Principles of Computer System Design

An Introduction

Chapter 9
Atomicity: All-or-Nothing and Before-or-After

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计算机系统设计原理

An Introduction

第9章 原子性：全有或全无与先后关系

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Overview

This chapter explores two closely related system engineering design strategies. The first is *all-or-nothing atomicity*, a design strategy for masking failures that occur while interpreting programs. The second is *before-or-after atomicity*, a design strategy for coordinating concurrent activities. Chapter 8[on-line] introduced failure masking, but did not show how to mask failures of running programs. Chapter 5 introduced coordination of concurrent activities, and presented solutions to several specific problems, but it did not explain any systematic way to ensure that actions have the before-or-after property. This chapter explores ways to systematically synthesize a design that provides both the all-or-nothing property needed for failure masking and the before-or-after property needed for coordination.

Many useful applications can benefit from atomicity. For example, suppose that you are trying to buy a toaster from an Internet store. You click on the button that says "purchase", but before you receive a response the power fails. You would like to have some assurance that, despite the power failure, either the purchase went through properly or that nothing happen at all. You don't want to find out later that your credit card was charged but the Internet store didn't receive word that it was supposed to ship the toaster. In other words, you would like to see that the action initiated by the "purchase" button be all-or-nothing despite the possibility of failure. And if the store has only one toaster in stock and two customers both click on the "purchase" button for a toaster at about the same time, one of the customers should receive a confirmation of the purchase, and the other should receive a "sorry, out of stock" notice. It would be problematic if

概述

本章探讨两种紧密相关的系统工程设计策略。第一种是*all-or-nothing atomicity*，一种用于掩盖程序解释过程中发生故障的设计策略。第二种是*before-or-after atomicity*，一种用于协调并发活动的设计策略。第8章[在线]介绍了故障掩盖，但未展示如何掩盖运行中程序的故障。第5章介绍了并发活动的协调，并针对几个具体问题提供了解决方案，但未阐明确保操作具备前后一致性的系统方法。本章将探索如何系统性地综合出一种设计，既能提供故障掩盖所需的全有或全无属性，又能满足协调所需的前后一致性属性。

许多实用的应用都能从原子性中受益。例如，假设你正尝试从一家网店购买一台烤面包机。你点击了标有“购买”的按钮，但在收到响应之前断电了。你希望得到某种保证，即尽管发生了断电，要么购买已正确完成，要么什么都没发生。你不希望事后发现信用卡被扣款，而网店却未收到发货烤面包机的通知。换句话说，你希望看到由“购买”按钮发起的动作，即便在可能出现故障的情况下，也能保持“全有或全无”的特性。如果商店库存中仅剩一台烤面包机，而两位顾客几乎同时点击了“购买”按钮，那么其中一位顾客应收到购买确认，另一位则应收到“抱歉，缺货”的通知。如果出现其他情况，就会有问题。

both customers received confirmations of purchase. In other words, both customers would like to see that the activity initiated by their own click of the "purchase" button occur either completely before or completely after any other, concurrent click of a "purchase" button.

The single conceptual framework of atomicity provides a powerful way of thinking about both all-or-nothing failure masking and before-or-after sequencing of concurrent activities. *Atomicity* is the performing of a sequence of steps, called *actions*, so that they appear to be done as a single, indivisible step, known in operating system and architecture literature as an *atomic action* and in database management literature as a *transaction*. When a fault causes a failure in the middle of a correctly designed atomic action, it will appear to the invoker of the atomic action that the atomic action either completed successfully or did nothing at all—thus an atomic action provides all-or-nothing atomicity. Similarly, when several atomic actions are going on concurrently, each atomic action will appear to take place either completely before or completely after every other atomic action—thus an atomic action provides before-or-after atomicity. Together, all-or-nothing atomicity and before-or-after atomicity provide a particularly strong form of modularity: they hide the fact that the atomic action is actually composed of multiple steps.

The result is a *sweeping simplification* in the description of the possible states of a system. This simplification provides the basis for a methodical approach to recovery from failures and coordination of concurrent activities that simplifies design, simplifies understanding for later maintainers, and simplifies verification of correctness. These desiderata are particularly important because errors caused by mistakes in coordination usually depend on the relative timing of external events and among different threads. When a timing-dependent error occurs, the difficulty of discovering and diagnosing it can be orders of magnitude greater than that of finding a mistake in a purely sequential activity. The reason is that even a small number of concurrent activities can have a very large number of potential real time sequences. It is usually impossible to determine which of those many potential sequences of steps preceded the error, so it is effectively impossible to reproduce the error under more carefully controlled circumstances. Since debugging this class of error is so hard, techniques that ensure correct coordination *a priori* are particularly valuable.

The remarkable thing is that the same systematic approach—atomicity—to failure recovery also applies to coordination of concurrent activities. In fact, since one must be able to deal with failures while at the same time coordinating concurrent activities, any attempt to use different strategies for these two problems requires that the strategies be compatible. Being able to use the same strategy for both is another *sweeping simplification*.

Atomic actions are a fundamental building block that is widely applicable in computer system design. Atomic actions are found in database management systems, in register management for pipelined processors, in file systems, in change-control systems used for program development, and in many everyday applications such as word processors and calendar managers.

两位客户都收到了购买确认。换句话说，两位客户都希望看到，由他们自己点击“购买”按钮所发起的活动，要么完全在其他并发“购买”按钮点击之前完成，要么完全在其之后完成。

原子性的单一概念框架为思考全有或全无的故障屏蔽以及并发活动的前后排序提供了一种强有力的方式。*Atomicity*是执行一系列被称为*actions*的步骤，使其看起来像是完成了一个单一、不可分割的步骤，这在操作系统和架构文献中被称为*atomic action*，在数据库管理文献中则称为*transaction*。当故障在一个正确设计的原子动作中间导致失败时，对于原子动作的调用者来说，该动作要么成功完成，要么什么都没做——因此原子动作提供了全有或全无的原子性。类似地，当多个原子动作并发进行时，每个原子动作看起来要么完全在其他所有原子动作之前发生，要么完全在其之后——因此原子动作提供了前后原子性。全有或全无的原子性与前后原子性共同提供了一种特别强大的模块化形式：它们隐藏了原子动作实际上由多个步骤组成的事实。

结果是在系统可能状态的描述中出现了一个*sweeping simplification*。这一简化为从故障中恢复和并发活动的协调提供了系统化的方法基础，它简化了设计，便于后续维护者的理解，也简化了正确性的验证。这些需求尤为重要，因为由协调错误引发的故障通常依赖于外部事件及不同线程间的相对时序。当时序相关的错误发生时，发现和诊断它的难度可能比在纯顺序活动中找出错误高出数个数量级。原因在于，即便是少量的并发活动也可能产生极其庞大的实时序列组合。通常无法确定在众多可能的步骤序列中，究竟是哪一种导致了错误，因此实际上难以在更受控的环境中复现该错误。鉴于调试此类错误的难度极高，确保正确协调*a priori*的技术显得尤为宝贵。

值得注意的是，同样的系统性方法——原子性——不仅适用于故障恢复，也适用于并发活动的协调。事实上，由于必须在协调并发活动的同时处理故障，若试图对这两个问题采用不同策略，就必须确保这些策略相互兼容。能够对两者使用同一策略，则是另一个*sweeping simplification*优势。

原子操作是计算机系统设计中广泛应用的基本构建模块。在数据库管理系统、流水线处理器的寄存器管理、文件系统、用于程序开发的变更控制系统，以及诸如文字处理器和日历管理器等众多日常应用中，都能见到原子操作的身影。

Sidebar 9.1: Actions and transactions The terminology used by system designers to discuss atomicity can be confusing because the concept was identified and developed independently by database designers and by hardware architects.

An action that changes several data values can have any or all of at least four independent properties: it can be *all-or-nothing* (either all or none of the changes happen), it can be *before-or-after* (the changes all happen either before or after every concurrent action), it can be *constraint-maintaining* (the changes maintain some specified invariant), and it can be *durable* (the changes last as long as they are needed).

Designers of database management systems customarily are concerned only with actions that are both all-or-nothing and before-or-after, and they describe such actions as *transactions*. In addition, they use the term *atomic* primarily in reference to all-or-nothing atomicity. On the other hand, hardware processor architects customarily use the term *atomic* to describe an action that exhibits before-or-after atomicity.

This book does not attempt to change these common usages. Instead, it uses the qualified terms "all-or-nothing atomicity" and "before-or-after atomicity." The unqualified term "atomic" may imply all-or-nothing, or before-or-after, or both, depending on the context. The text uses the term "transaction" to mean an action that is *both* all-or-nothing and before-or-after.

All-or-nothing atomicity and before-or-after atomicity are universally defined properties of actions, while constraints are properties that different applications define in different ways. Durability lies somewhere in between because different applications have different durability requirements. At the same time, implementations of constraints and durability usually have a prerequisite of atomicity. Since the atomicity properties are modularly separable from the other two, this chapter focuses just on atomicity. Chapter 10[on-line] then explores how a designer can use transactions to implement constraints and enhance durability.

The sections of this chapter define atomicity, examine some examples of atomic actions, and explore systematic ways of achieving atomicity: *version histories*, *logging*, and *locking protocols*. Chapter 10[on-line] then explores some applications of atomicity. Case studies at the end of both chapters provide real-world examples of atomicity as a tool for creating useful systems.

9.1 Atomicity

Atomicity is a property required in several different areas of computer system design. These areas include managing a database, developing a hardware architecture, specifying the interface to an operating system, and more generally in software engineering. The table below suggests some of the kinds of problems to which atomicity is applicable. In

边栏9.1：操作与事务　系统设计者在讨论原子性时使用的术语可能会令人困惑，因为这一概念是由数据库设计者和硬件架构师各自独立发现并发展的。

一个改变多个数据值的动作可以具备以下至少四种独立属性中的任意一种或全部：它可以是*all-or-nothing*（要么所有更改都发生，要么都不发生），可以是*before-or-after*（所有更改都在每个并发动作之前或之后完成），可以是*constraint-maintaining*（这些更改保持某些特定的不变性），还可以是*durable*（更改持续到不再需要为止）。

数据库管理系统的设计者通常只关注那些要么全有要么全无且具备前后一致性的操作，并将此类操作描述为*transactions*。此外，他们使用术语*atomic*主要指的是全有或全无的原子性。另一方面，硬件处理器架构师习惯上用术语*atomic*来描述展现前后一致性原子性的操作。

本书并不试图改变这些常见用法。相反，它使用了限定性术语“全有或全无原子性”和“之前或之后原子性”。不加限定的“原子”一词可能隐含全有或全无、之前或之后，或两者兼具，具体取决于上下文。文中使用“事务”一词来表示一个*both*全有或全无且之前或之后的行为。

全有或全无原子性与前后原子性是动作普遍定义的属性，而约束则是不同应用以不同方式定义的属性。持久性介于两者之间，因为不同应用对持久性有着不同的要求。与此同时，约束和持久性的实现通常以原子性为前提。由于原子性属性可与其他两者模块化分离，本章仅聚焦于原子性。第10章[在线]将探讨设计者如何利用事务来实现约束并增强持久性。

本章各节将定义原子性，考察一些原子操作的实例，并探讨实现原子性的系统化方法：*version histories*、*logging*和*locking protocols*。随后，第10章[在线]将探讨原子性的一些应用。两章末尾的案例研究提供了原子性作为构建实用系统工具的真实世界示例。

9.1 原子性

原子性是计算机系统设计中多个不同领域所需的一种属性。这些领域包括数据库管理、硬件架构开发、操作系统接口规范，以及更广泛的软件工程领域。下表列举了原子性适用的一些问题类型。在

this chapter we will encounter examples of both kinds of atomicity in each of these different areas.

Area	All-or-nothing atomicity	Before-or-after atomicity
database management	updating more than one record	records shared between threads
hardware architecture	handling interrupts and exceptions	register renaming
operating systems	supervisor call interface	printer queue
software engineering	handling faults in layers	bounded buffer

9.1.1 All-or-Nothing Atomicity in a Database

As a first example, consider a database of bank accounts. We define a procedure named TRANSFER that debits one account and credits a second account, both of which are stored on disk, as follows:

```
1    procedure TRANSFER (debit_account, credit_account, amount)
2        GET (dbdata, debit_account)
3        dbdata ← dbdata - amount
4        PUT (dbdata, debit_account)
5        GET (crdata, credit_account)
6        crdata ← crdata + amount
7        PUT (crdata, credit_account)
```

where *debit_account* and *credit_account* identify the records for the accounts to be debited and credited, respectively.

Suppose that the system crashes while executing the PUT instruction on line 4. Even if we use the MORE_DURABLE_PUT described in Section 8.5.4, a system crash at just the wrong time may cause the data written to the disk to be scrambled, and the value of *debit_account* lost. We would prefer that either the data be completely written to the disk or nothing be written at all. That is, we want the PUT instruction to have the all-or-nothing atomicity property. Section 9.2.1 will describe a way to do that.

There is a further all-or-nothing atomicity requirement in the TRANSFER procedure. Suppose that the PUT on line 4 is successful but that while executing line 5 or line 6 the power fails, stopping the computer in its tracks. When power is restored, the computer restarts, but volatile memory, including the state of the thread that was running the TRANSFER procedure, has been lost. If someone now inquires about the balances in *debit_account* and in *credit_account* things will not add up properly because *debit_account* has a new value but *credit_account* has an old value. One might suggest postponing the first PUT to be just before the second one, but that just reduces the window of vulnerability, it does not eliminate it—the power could still fail in between the two PUTs. To eliminate the window, we must somehow arrange that the two PUT instructions, or perhaps even the entire TRANSFER procedure, be done as an all-or-nothing atomic

本章我们将在这些不同的领域中遇到两种原子性的实例。

Area	All-or-nothing atomicity	Before-or-after atomicity
database management	updating more than one record	records shared between threads
hardware architecture	handling interrupts and exceptions	register renaming
operating systems	supervisor call interface	printer queue
software engineering	handling faults in layers	bounded buffer

9.1.1 数据库中的全有或全无原子性

作为第一个例子，考虑一个银行账户的数据库。我们定义一个名为TRANSFER的过程，该过程从一个账户扣款并存入第二个账户，这两个账户都存储在磁盘上，具体如下：

```
1       过程 转移 (debit_account, credit_account, amount) 2  获取
(dbdata, debit_account) - amount  放置 (dbdata, debit_account) 5  获取
3       dbdata ← dbdata + amount  放置 (crdata, credit_account)
6       crdata ← crdata
```

其中*debit_account*和*credit_account* 分别标识待借记和贷记的账户记录。

假设系统在执行第4行的PUT指令时崩溃。即便我们采用了8.5.4节所述的MORE_DURABLE_PUT方法，在极其不巧的时刻发生系统崩溃仍可能导致写入磁盘的数据混乱，并丢失*debit_account*的值。我们更希望数据要么完整写入磁盘，要么完全不写入。换言之，我们希望PUT指令具备全有或全无的原子性特性。9.2.1节将阐述实现这一目标的方法。

在TRANSFER过程中还存在一个全有或全无的原子性要求。假设第4行的PUT操作成功，但在执行第5行或第6行时电源中断，导致计算机突然停止。当电源恢复后，计算机重启，但包括运行TRANSFER过程的线程状态在内的易失性内存已经丢失。此时若有人查询*debit_account*和*credit_account*中的余额，数值将无法正确匹配，因为*debit_account* 已更新为新值而*credit_account*仍保留旧值。有人可能建议将第一个PUT操作推迟至第二个操作之前执行，但这只是缩小了风险窗口而非彻底消除——电源仍可能在两个PUT操作之间发生故障。要消除这个风险窗口，我们必须设法确保这两个PUT指令，甚至整个TRANSFER过程，以全有或全无的原子方式完成。

action. In Section 9.2.3 we will devise a TRANSFER procedure that has the all-or-nothing property, and in Section 9.3 we will see some additional ways of providing the property.

9.1.2 All-or-Nothing Atomicity in the Interrupt Interface

A second application for all-or-nothing atomicity is in the processor instruction set interface as seen by a thread. Recall from Chapters 2 and 5 that a thread normally performs actions one after another, as directed by the instructions of the current program, but that certain events may catch the attention of the thread's interpreter, causing the interpreter, rather than the program, to supply the next instruction. When such an event happens, a different program, running in an interrupt thread, takes control.

If the event is a signal arriving from outside the interpreter, the interrupt thread may simply invoke a thread management primitive such as ADVANCE, as described in Section 5.6.4, to alert some other thread about the event. For example, an I/O operation that the other thread was waiting for may now have completed. The interrupt handler then returns control to the interrupted thread. This example requires before-or-after atomicity between the interrupt thread and the interrupted thread. If the interrupted thread was in the midst of a call to the thread manager, the invocation of ADVANCE by the interrupt thread should occur either before or after that call.

Another possibility is that the interpreter has detected that something is going wrong in the interrupted thread. In that case, the interrupt event invokes an exception handler, which runs in the environment of the original thread. (Sidebar 9.2 offers some examples.) The exception handler either adjusts the environment to eliminate some problem (such as a missing page) so that the original thread can continue, or it declares that the original thread has failed and terminates it. In either case, the exception handler will need to examine the state of the action that the original thread was performing at the instant of the interruption—was that action finished, or is it in a partially done state?

Ideally, the handler would like to see an all-or-nothing report of the state: either the instruction that caused the exception completed or it didn't do anything. An all-or-nothing report means that the state of the original thread is described entirely with values belonging to the layer in which the exception handler runs. An example of such a value is the program counter, which identifies the next instruction that the thread is to execute. An in-the-middle report would mean that the state description involves values of a lower layer, probably the operating system or the hardware processor itself. In that case, knowing the next instruction is only part of the story; the handler would also need to know which parts of the current instruction were executed and which were not. An example might be an instruction that increments an address register, retrieves the data at that new address, and adds that data value to the value in another register. If retrieving the data causes a missing-page exception, the description of the current state is that the address register has been incremented but the retrieval and addition have not yet been performed. Such an in-the-middle report is problematic because after the handler retrieves the missing page it cannot simply tell the processor to jump to the instruction that failed—that would increment the address register again, which is not what the program-

动作。在9.2.3节中，我们将设计一个具有全有或全无特性的TRANSFER 程序，而在9.3节中，我们将看到提供该特性的一些其他方法。

9.1.2 中断接口中的全有或全无原子性

全有或全无原子性的第二个应用体现在线程所见的处理器指令集接口中。回顾第2章和第5章的内容，线程通常按照当前程序的指令依次执行操作，但某些事件可能会引起线程解释器的注意，导致由解释器而非程序提供下一条指令。当此类事件发生时，运行在中断线程中的另一个程序将接管控制权。

如果事件是从解释器外部到达的信号，中断线程可以简单地调用一个线程管理原语，如ADVANCE，如第5.6.4节所述，以向其他线程通报该事件。例如，其他线程正在等待的I/O操作可能现已完成。随后，中断处理程序将控制权交还给被中断的线程。此例要求中断线程与被中断线程之间具备前或后原子性。若被中断线程当时正在调用线程管理器，则中断线程对ADVANCE的调用应当发生在该调用之前或之后。

另一种可能是解释器检测到被中断线程中出现了问题。此时，中断事件会调用一个异常处理器，该处理器在原始线程的环境中运行。（侧边栏9.2提供了一些示例。）异常处理器要么调整环境以消除某些问题（如缺页错误），使原始线程得以继续执行；要么判定原始线程失败并将其终止。无论哪种情况，异常处理器都需要检查原始线程在中断瞬间所执行操作的状态——该操作是已完成，还是处于部分完成状态?

理想情况下，异常处理程序希望看到关于状态的“全有或全无”报告：要么引发异常的指令已完成，要么它未执行任何操作。全有或全无报告意味着原始线程的状态完全由异常处理程序所在层的值来描述。此类值的一个例子是程序计数器，它标识线程接下来要执行的指令。而中间态报告则意味着状态描述涉及更低层的值，可能是操作系统或硬件处理器本身的值。在这种情况下，仅知道下一条指令是不够的；处理程序还需要了解当前指令的哪些部分已执行、哪些部分未执行。例如，一条指令可能先递增地址寄存器，再从新地址处获取数据，最后将该数据值与另一寄存器中的值相加。若获取数据时触发缺页异常，则当前状态的描述为地址寄存器已递增，但数据获取与加法操作尚未执行。这种中间态报告会带来问题，因为处理程序在获取缺失页面后，不能简单地让处理器跳转至失败的指令重新执行——这会导致地址寄存器再次递增，而程序的本意并非如此。

Sidebar 9.2: Events that might lead to invoking an exception handler

1. A hardware fault occurs:
   - The processor detects a memory parity fault.
   - A sensor reports that the electric power has failed; the energy left in the power supply may be just enough to perform a graceful shutdown.
2. A hardware or software interpreter encounters something in the program that is clearly wrong:
   - The program tried to divide by zero.
   - The program supplied a negative argument to a square root function.
3. Continuing requires some resource allocation or deferred initialization:
   - The running thread encountered a missing-page exception in a virtual memory system.
   - The running thread encountered an indirection exception, indicating that it encountered an unresolved procedure linkage in the current program.
4. More urgent work needs to take priority, so the user wishes to terminate the thread:
   - This program is running much longer than expected.
   - The program is running normally, but the user suddenly realizes that it is time to catch the last train home.
5. The user realizes that something is wrong and decides to terminate the thread:
   - Calculating *e*, the program starts to display 3.1415…
   - The user asked the program to copy the wrong set of files.
6. Deadlock:
   - Thread A has acquired the scanner, and is waiting for memory to become free; thread B has acquired all available memory, and is waiting for the scanner to be released. Either the system notices that this set of waits cannot be resolved or, more likely, a timer that should never expire eventually expires. The system or the timer signals an exception to one or both of the deadlocked threads.

mer expected. Jumping to the next instruction isn't right, either, because that would omit the addition step. An all-or-nothing report is preferable because it avoids the need for the handler to peer into the details of the next lower layer. Modern processor designers are generally careful to avoid designing instructions that don't have the all-or-nothing property. As will be seen shortly, designers of higher-layer interpreters must be similarly careful.

Sections 9.1.3 and 9.1.4 explore the case in which the exception terminates the running thread, thus creating a fault. Section 9.1.5 examines the case in which the interrupted thread continues, oblivious (one hopes) to the interruption.

侧边栏 9.2：可能导致调用异常处理程序的事件

1. 发生硬件故障：
   - •处理器检测到内存奇偶校验错误。
   - •传感器报告电力已中断；电源中剩余的能量可能仅够执行一次正常关机。

2. 硬件或软件解释器在程序中遇到明显错误的内容：
   - •程序试图除以零。
   - •程序向平方根函数提供了一个负数参数。

3. 继续需要一些资源分配或延迟初始化：
   - •运行线程在虚拟内存系统中遇到了缺页异常。
   - •运行线程遇到了一个间接异常，表明它在当前程序中遇到了未解析的过程链接。

4. 更紧急的工作需要优先处理，因此用户希望终止线程：
   - •这个程序的运行时间远超预期。
   - •程序运行正常，但用户突然意识到该赶最后一班火车回家了。

5. 用户意识到出了问题，决定终止线程：
   - •计算π时，程序开始显示3.1415…
   - •用户要求程序复制了错误的文件集。

6. 死锁：
   - •线程A已获取扫描器，并正等待内存释放；线程B则占用了所有可用内存，同时等待扫描器被释放。系统要么会检测到这一组等待状态无法解除，要么更常见的是，某个本不应触发的计时器最终超时。系统或计时器会向一个或两个死锁线程发出异常信号。

正如预期的那样。直接跳转到下一条指令也不正确，因为这会遗漏加法步骤。全有或全无的报告更为可取，因为它避免了处理程序需要窥探下一层细节的情况。现代处理器设计者通常会谨慎避免设计不具备全有或全无特性的指令。很快我们将看到，高层解释器的设计者也必须同样谨慎。

9.1.3节和9.1.4节探讨了异常导致运行线程终止，从而引发故障的情况。9.1.5节则分析了被中断线程继续执行（但愿）未察觉中断的情形。

9.1.3 All-or-Nothing Atomicity in a Layered Application

A third example of all-or-nothing atomicity lies in the challenge presented by a fault in a running program: at the instant of the fault, the program is typically in the middle of doing something, and it is usually not acceptable to leave things half-done. Our goal is to obtain a more graceful response, and the method will be to require that some sequence of actions behave as an atomic action with the all-or-nothing property. Atomic actions are closely related to the modularity that arises when things are organized in layers. Layered components have the feature that a higher layer can completely hide the existence of a lower layer. This hiding feature makes layers exceptionally effective at error containment and for systematically responding to faults.

To see why, recall the layered structure of the calendar management program of Chapter 2, reproduced in Figure 9.19.1 (that figure may seem familiar—it is a copy of Figure 2.10). The calendar program implements each request of the user by executing a sequence of Java language statements. Ideally, the user will never notice any evidence of the composite nature of the actions implemented by the calendar manager. Similarly, each statement of the Java language is implemented by several actions at the hardware layer. Again, if the Java interpreter is carefully implemented, the composite nature of the implementation in terms of machine language will be completely hidden from the Java programmer.

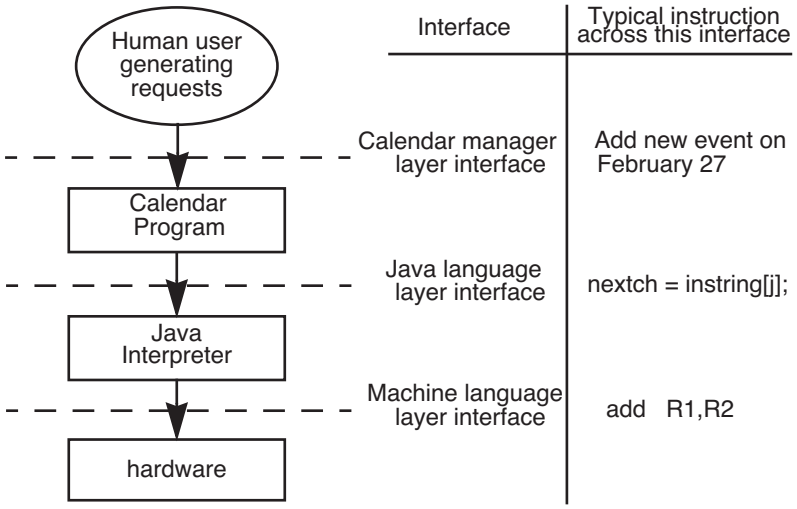

FIGURE 9.1

An application system with three layers of interpretation. The user has requested an action that will fail, but the failure will be discovered at the lowest layer. A graceful response involves atomicity at each interface.

9.1.3 分层应用中的全有或全无原子性

第三个全有或全无原子性的例子体现在运行程序出现故障时的挑战：在故障发生的瞬间，程序通常正处于执行某项操作的过程中，而让操作半途而废通常是不可接受的。我们的目标是获得更优雅的响应方式，方法则是要求某些动作序列必须表现为具有全有或全无特性的原子动作。原子动作与分层组织所呈现的模块性密切相关。分层组件的特点是高层能够完全隐藏底层的存在。这种隐藏特性使分层结构在错误隔离和系统化应对故障方面表现出色。

要理解这一点，可以回顾第2章日历管理程序的分层结构，如图9.19.1所示（该图可能看起来很熟悉——它是图2.10的副本）。日历程序通过执行一系列Java语言语句来实现用户的每个请求。理想情况下，用户永远不会察觉到日历管理器所实现动作的复合性质。同样，Java语言的每条语句都由硬件层的多个动作实现。如果Java解释器被精心实现，那么基于机器语言的实现细节将完全对Java程序员隐藏。

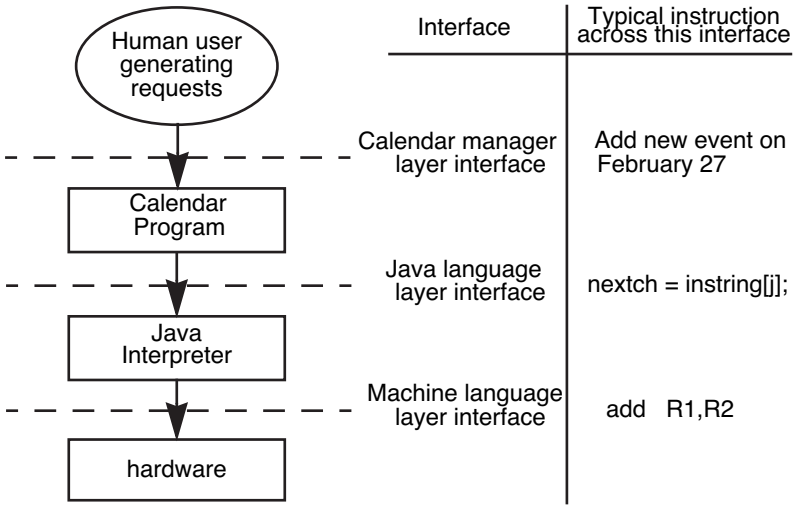

图9.1

一个具有三层解释的应用系统。用户请求的操作将会失败，但失败将在最底层被发现。优雅的响应需要在每个接口层面保持原子性{v*}。

Now consider what happens if the hardware processor detects a condition that should be handled as an exception—for example, a register overflow. The machine is in the middle of interpreting an action at the machine language layer interface—an ADD instruction somewhere in the middle of the Java interpreter program. That ADD instruction is itself in the middle of interpreting an action at the Java language interface—a Java expression to scan an array. That Java expression in turn is in the middle of interpreting an action at the user interface—a request from the user to add a new event to the calendar. The report "Overflow exception caused by the ADD instruction at location 41574" is not intelligible to the user at the user interface; that description is meaningful only at the machine language interface. Unfortunately, the implication of being "in the middle" of higher-layer actions is that the only accurate description of the current state of affairs is in terms of the progress of the machine language program.

The actual state of affairs in our example as understood by an all-seeing observer might be the following: the register overflow was caused by adding one to a register that contained a two's complement negative one at the machine language layer. That machine language add instruction was part of an action to scan an array of characters at the Java layer and a zero means that the scan has reached the end of the array. The array scan was embarked upon by the Java layer in response to the user's request to add an event on February 31. The highest-level interpretation of the overflow exception is "You tried to add an event on a non-existent date". We want to make sure that this report goes to the end user, rather than the one about register overflow. In addition, we want to be able to assure the user that this mistake has not caused an empty event to be added somewhere else in the calendar or otherwise led to any other changes to the calendar. Since the system couldn't do the requested change it should do nothing but report the error. Either a low-level error report or muddled data would reveal to the user that the action was composite.

With the insight that in a layered application, we want a fault detected by a lower layer to be contained in a particular way we can now propose a more formal definition of all-or-nothing atomicity:

All-or-nothing atomicity

A sequence of steps is an *all-or-nothing action* if, from the point of view of its invoker, the sequence always either

- ***completes*,**

or

- **aborts in such a way that it appears that the sequence had never been undertaken in the first place. That is, it *backs out*.**

In a layered application, the idea is to design each of the actions of each layer to be all-or-nothing. That is, whenever an action of a layer is carried out by a sequence of

现在考虑如果硬件处理器检测到一个应作为异常处理的情况会发生什么——例如，寄存器溢出。机器正处于机器语言层接口解释某个动作的过程中——即Java解释器程序中的某个ADD指令。这条ADD指令本身又处于解释Java语言接口动作的过程中——一个用于扫描数组的Java表达式。而这个Java表达式反过来又处于解释用户界面动作的过程中——用户向日历添加新事件的请求。报告“由位置41574处的ADD指令引起的溢出异常”对用户界面的用户来说是不可理解的；该描述仅在机器语言接口层面才有意义。遗憾的是，处于高层动作“中间”状态的隐含意义在于，对当前事态唯一准确的描述只能通过机器语言程序的执行进度来体现。

在我们这个例子中，一位全知观察者所理解的实际状况可能是这样的：寄存器溢出是由于在机器语言层面对一个存储着二进制补码负一的寄存器进行了加一操作。这条机器语言加法指令，在Java层面对应的是扫描字符数组的动作，而零值意味着扫描已到达数组末尾。Java层发起数组扫描，是为了响应用户要求在2月31日添加事件的请求。对于溢出异常的最高层级解读是"您试图在一个不存在的日期添加事件"。我们希望确保最终用户收到的是这份报告，而非关于寄存器溢出的技术细节。此外，我们还需要向用户保证：这个错误既不会在日历的其他位置生成空事件，也不会导致日历数据发生任何其他变更。由于系统无法执行请求的修改，它应当仅报告错误而不进行任何操作。无论是低层级的错误报告还是混乱的数据，都会向用户暴露出该动作的复合性。

基于分层应用的洞察，我们希望由下层检测到的故障能以特定方式被隔离，由此可以提出一个更形式化的全有或全无原子性定义：

全有或全无原子性

一系列步骤构成一个***all-or-nothing action***，如果从调用者的角度来看，该序列总是要么

- ***completes***,

或

- 以这样一种方式中止，使得看起来这个序列从未被首先执行过。也就是说，它***backs out.***

在分层应用程序中，设计理念是让每一层的每个动作都遵循“全有或全无”原则。也就是说，每当某一层的某个动作通过一系列操作执行时，

actions of the next lower layer, the action either completes what it was asked to do or else it backs out, acting as though it had not been invoked at all. When control returns to a higher layer after a lower layer detects a fault, the problem of being "in the middle" of an action thus disappears.

In our calendar management example, we might expect that the machine language layer would complete the add instruction but signal an overflow exception; the Java interpreter layer would, upon receiving the overflow exception might then decide that its array scan has ended, and return a report of "scan complete, value not found" to the calendar management layer; the calendar manager would take this not-found report as an indication that it should back up, completely undo any tentative changes, and tell the user that the request to add an event on that date could not be accomplished because the date does not exist.

Thus some layers run to completion, while others back out and act as though they had never been invoked, but either way the actions are all-or-nothing. In this example, the failure would probably propagate all the way back to the human user to decide what to do next. A different failure (e.g. "there is no room in the calendar for another event") might be intercepted by some intermediate layer that knows of a way to mask it (e.g., by allocating more storage space). In that case, the all-or-nothing requirement is that the layer that masks the failure find that the layer below has either never started what was to be the current action or else it has completed the current action but has not yet undertaken the next one.

All-or-nothing atomicity is not usually achieved casually, but rather by careful design and specification. Designers often get it wrong. An unintelligible error message is the typical symptom that a designer got it wrong. To gain some insight into what is involved, let us examine some examples.

9.1.4 Some Actions With and Without the All-or-Nothing Property

Actions that lack the all-or-nothing property have frequently been discovered upon adding multilevel memory management to a computer architecture, especially to a processor that is highly pipelined. In this case, the interface that needs to be all-or-nothing lies between the processor and the operating system. Unless the original machine architect designed the instruction set with missing-page exceptions in mind, there may be cases in which a missing-page exception can occur "in the middle" of an instruction, after the processor has overwritten some register or after later instructions have entered the pipeline. When such a situation arises, the later designer who is trying to add the multilevel memory feature is trapped. The instruction cannot run to the end because one of the operands it needs is not in real memory. While the missing page is being retrieved from secondary storage, the designer would like to allow the operating system to use the processor for something else (perhaps even to run the program that fetches the missing page), but reusing the processor requires saving the state of the currently executing program, so that it can be restarted later when the missing page is available. The problem is how to save the next-instruction pointer.

下一层级的动作要么完成其被要求执行的任务，要么完全回退，表现得仿佛从未被调用过。当控制权在底层检测到故障后返回到更高层级时，处于动作“中间状态”的问题便随之消失。

在我们的日历管理示例中，我们可能预期机器语言层会完成加法指令但触发溢出异常；Java解释器层在接收到溢出异常后，可能会判定其数组扫描已结束，并向日历管理层返回“扫描完成，未找到值”的报告；日历管理器会将此未找到报告视为应回退的信号，彻底撤销所有暂存更改，并告知用户由于该日期不存在，无法完成添加该日事件的请求。

因此，某些层级会运行至完成，而其他层级则会回退，表现得仿佛从未被调用过。但无论哪种情况，操作都是全有或全无的。在这个例子中，失败可能会一直传递回人类用户，由其决定下一步行动。而另一种失败（例如“日历中没有空间安排更多事件”）可能会被某个中间层拦截，该层知道如何掩盖这一失败（比如通过分配更多存储空间）。在这种情况下，全有或全无的要求在于：掩盖失败的层级必须发现，下层要么从未启动本应成为当前操作的动作，要么已经完成当前操作但尚未着手下一个动作。

全有或全无的原子性通常不是偶然实现的，而是通过精心的设计和规范达成的。设计者常常会犯错，而难以理解的错误信息正是设计出错的典型表现。为了深入理解其中的关键所在，让我们来看几个例子。

9.1.4 具有与不具有全有或全无属性的某些行为

在计算机架构中，尤其是高度流水线化的处理器上添加多级内存管理时，经常发现某些操作缺乏“全有或全无”特性。此时，处理器与操作系统之间的接口必须满足这一特性。若原始机器架构师在设计指令集时未充分考虑缺页异常，就可能出现指令执行“中途”触发缺页异常的情况——这可能在处理器覆写某些寄存器后发生，也可能在后续指令已进入流水线后发生。面对这种困境，试图添加多级内存功能的后继设计者将陷入两难：指令无法执行到底，因其所需操作数不在实存中；而从二级存储提取缺失页时，设计者希望允许操作系统将处理器转作他用（甚至可能运行获取缺失页的程序），但重用处理器需保存当前执行程序的状态以便后续重启。核心难题在于如何保存下一条指令指针。

If every instruction is an all-or-nothing action, the operating system can simply save as the value of the next-instruction pointer the address of the instruction that encountered the missing page. The resulting saved state description shows that the program is between two instructions, one of which has been completely executed, and the next one of which has not yet begun. Later, when the page is available, the operating system can restart the program by reloading all of the registers and setting the program counter to the place indicated by the next-instruction pointer. The processor will continue, starting with the instruction that previously encountered the missing page exception; this time it should succeed. On the other hand, if even one instruction of the instruction set lacks the all-or-nothing property, when an interrupt happens to occur during the execution of that instruction it is not at all obvious how the operating system can save the processor state for a future restart. Designers have come up with several techniques to retrofit the all-or-nothing property at the machine language interface. Section 9.8 describes some examples of machine architectures that had this problem and the techniques that were used to add virtual memory to them.

A second example is the supervisor call (SVC). Section 5.3.4 pointed out that the SVC instruction, which changes both the program counter and the processor mode bit (and in systems with virtual memory, other registers such as the page map address register), needs to be all-or-nothing, to ensure that all (or none) of the intended registers change. Beyond that, the SVC invokes some complete kernel procedure. The designer would like to arrange that the entire call, (the combination of the SVC instruction and the operation of the kernel procedure itself) be an all-or-thing action. An all-or-nothing design allows the application programmer to view the kernel procedure as if it is an extension of the hardware. That goal is easier said than done, since the kernel procedure may detect some condition that prevents it from carrying out the intended action. Careful design of the kernel procedure is thus required.

Consider an SVC to a kernel READ procedure that delivers the next typed keystroke to the caller. The user may not have typed anything yet when the application program calls READ, so the the designer of READ must arrange to wait for the user to type something. By itself, this situation is not especially problematic, but it becomes more so when there is also a user-provided exception handler. Suppose, for example, a thread timer can expire during the call to READ and the user-provided exception handler is to decide whether or not the thread should continue to run a while longer. The scenario, then, is the user program calls READ, it is necessary to wait, and while waiting, the timer expires and control passes to the exception handler. Different systems choose one of three possibilities for the design of the READ procedure, the last one of which is not an all-or-nothing design:

1. *An all-or-nothing design that implements the "nothing" option (blocking read)*: Seeing no available input, the kernel procedure first adjusts return pointers ("push the PC back") to make it appear that the application program called AWAIT just ahead of its call to the kernel READ procedure and then it transfers control to the kernel AWAIT entry point. When the user finally types something, causing AWAIT to return, the user's thread re-executes the original kernel call to READ, this time finding the typed

如果每条指令都是一个“全有或全无”的操作，操作系统只需将遇到缺页异常的指令地址保存为下一条指令指针的值即可。这样保存的状态描述表明，程序正处于两条指令之间，其中一条已完全执行完毕，而下一条尚未开始。之后，当页面可用时，操作系统可以通过重新加载所有寄存器并将程序计数器设置为下一条指令指针所指示的位置来重启程序。处理器将从先前触发缺页异常的指令处继续执行，这次应该能够成功。　　另一方面，如果指令集中哪怕有一条指令不具备“全有或全无”特性，当执行该指令期间恰好发生中断时，操作系统如何保存处理器状态以供未来重启就完全不明朗了。设计者们提出了多种技术方案，在机器语言接口层面补全这一特性。第9.8节将描述一些曾存在此问题的机器架构实例，以及为它们添加虚拟内存所采用的技术。

第二个例子是监督程序调用（SVC）。5.3.4节指出，SVC指令会同时改变程序计数器和处理器模式位（在具有虚拟内存的系统中，还包括其他寄存器，如页映射地址寄存器），因此需要确保所有目标寄存器要么全部更新，要么完全不更新，即实现全有或全无的操作。此外，SVC会调用某个完整的内核过程。设计者希望将整个调用过程（即SVC指令与内核过程本身的组合）安排为一个全有或全无的动作。这种全有或全无的设计让应用程序员能够将内核过程视为硬件的延伸。然而，这一目标说来容易做来难，因为内核过程可能会检测到某些条件阻碍其执行预期操作。因此，必须精心设计内核过程。

考虑一个SVC（系统调用）到内核过程READ，该过程将下一个键入的击键传递给调用者。当应用程序调用READ时，用户可能尚未输入任何内容，因此READ的设计者必须安排等待用户输入。仅就这一点而言，情况并不特别成问题，但当还存在用户提供的异常处理程序时，情况就变得更为复杂。例如，假设在调用READ期间线程计时器可能到期，而用户提供的异常处理程序需要决定线程是否应继续运行一段时间。于是，场景是这样的：用户程序调用READ，需要等待，而在等待过程中，计时器到期，控制权转交给异常处理程序。不同系统为READ过程的设计选择了三种可能性之一，其中最后一种并非非此即彼的设计：

1. *An all-or-nothing design that implements the “nothing” option (blocking read)*：当发现没有可用输入时，内核过程首先调整返回指针（“将PC压回”），使得应用程序看起来像是在调用内核READ过程之前调用了AWAIT，随后将控制权转移至内核AWAIT 入口点。当用户最终键入内容导致AWAIT返回时，用户的线程会重新执行最初对READ的内核调用，此时便能找到已输入的内容。

input. With this design, if a timer exception occurs while waiting, when the exception handler investigates the current state of the thread it finds the answer "the application program is between instructions; its next instruction is a call to READ." This description is intelligible to a user-provided exception handler, and it allows that handler several options. One option is to continue the thread, meaning go ahead and execute the call to READ. If there is still no input, READ will again push the PC back and transfer control to AWAIT. Another option is for the handler to save this state description with a plan of restoring a future thread to this state at some later time.

2. *An all-or-nothing design that implements the "all" option (non-blocking read)*: Seeing no available input, the kernel immediately returns to the application program with a zero-length result, expecting that the program will look for and properly handle this case. The program would probably test the length of the result and if zero, call AWAIT itself or it might find something else to do instead. As with the previous design, this design ensures that at all times the user-provided timer exception handler will see a simple description of the current state of the thread—it is between two user program instructions. However, some care is needed to avoid a race between the call to AWAIT and the arrival of the next typed character.

3. *A blocking read design that is neither "all" nor "nothing" and therefore not atomic*: The kernel READ procedure itself calls AWAIT, blocking the thread until the user types a character. Although this design seems conceptually simple, the description of the state of the thread from the point of view of the timer exception handler is not simple. Rather than "between two user instructions", it is "waiting for something to happen in the middle of a user call to kernel procedure READ". The option of saving this state description for future use has been foreclosed. To start another thread with this state description, the exception handler would need to be able to request "start this thread just after the call to AWAIT in the middle of the kernel READ entry." But allowing that kind of request would compromise the modularity of the user-kernel interface. The user-provided exception handler could equally well make a request to restart the thread anywhere in the kernel, thus bypassing its gates and compromising its security.

The first and second designs correspond directly to the two options in the definition of an all-or-nothing action, and indeed some operating systems offer both options. In the first design the kernel program acts in a way that appears that the call had never taken place, while in the second design the kernel program runs to completion every time it is called. Both designs make the kernel procedure an all-or-nothing action, and both lead to a user-intelligible state description—the program is between two of its instructions—if an exception should happen while waiting.

One of the appeals of the client/server model introduced in Chapter 4 is that it tends to force the all-or-nothing property out onto the design table. Because servers can fail independently of clients, it is necessary for the client to think through a plan for recovery

输入。采用这种设计后，如果在等待期间发生定时器异常，当异常处理程序检查线程的当前状态时，它会发现答案是“应用程序正处于指令之间；其下一条指令是调用READ。”这一描述对用户提供的异常处理程序而言是可理解的，并且为该处理程序提供了几种选择。一种选择是继续执行线程，即继续执行对READ的调用。如果此时仍无输入，READ会再次将程序计数器（PC）回推并将控制权转移给AWAIT。另一种选择是让处理程序保存这一状态描述，计划在未来的某个时间点将某个线程恢复至此状态。

2. *An all-or-nothing design that implements the “all” option (non-blocking read)*：当内核发现没有可用输入时，会立即向应用程序返回一个零长度的结果，预期程序会检测并妥善处理这种情况。程序可能会检查结果的长度，若为零，则自行调用AWAIT，或转而执行其他操作。与之前的设计相同，此设计确保用户提供的定时器异常处理程序在任何时候都能看到线程当前状态的简单描述——即线程处于两条用户程序指令之间。然而，需注意避免对AWAIT的调用与下一个键入字符到达之间发生竞态条件。

3. *A blocking read design that is neither “all” nor “nothing” and therefore not atomic*：内核READ过程本身调用了AWAIT，阻塞线程直至用户输入一个字符。尽管这一设计在概念上看似简单，但从定时器异常处理器的视角来描述线程的状态却并不简单。它并非“处于两条用户指令之间”，而是“在用户调用内核过程READ的过程中等待某事件发生”。保存这一状态描述以供后续使用的选项已被排除。若要以该状态描述启动另一线程，异常处理器需能够请求“在调用内核READ 入口中的AWAIT之后立即启动此线程”。但允许此类请求会损害用户-内核接口的模块性。用户提供的异常处理器同样可以请求在内核的任何位置重启线程，从而绕过其门控机制，危及安全性。

第一和第二种设计直接对应于全有或全无操作定义中的两种选项，事实上某些操作系统确实同时提供了这两种选择。在第一种设计中，内核程序的运行方式使得调用看似从未发生过；而在第二种设计中，每次调用内核程序时都会确保其完整执行完毕。这两种设计都将内核过程转变为全有或全无操作，并且如果在等待期间发生异常，两者都会产生用户可理解的状态描述——程序正处于其两条指令之间的状态。

第4章介绍的客户端/服务器模型的一大吸引力在于，它倾向于迫使设计直面"全有或全无"的特性。由于服务器可能独立于客户端发生故障，客户端必须周密考虑恢复方案。

from server failure, and a natural model to use is to make every action offered by a server all-or-nothing.

9.1.5 Before-or-After Atomicity: Coordinating Concurrent Threads

In Chapter 5 we learned how to express opportunities for concurrency by creating threads, the goal of concurrency being to improve performance by running several things at the same time. Moreover, Section 9.1.2 above pointed out that interrupts can also create concurrency. Concurrent threads do not represent any special problem until their paths cross. The way that paths cross can always be described in terms of shared, writable data: concurrent threads happen to take an interest in the same piece of writable data at about the same time. It is not even necessary that the concurrent threads be running simultaneously; if one is stalled (perhaps because of an interrupt) in the middle of an action, a different, running thread can take an interest in the data that the stalled thread was, and will sometime again be, working with.

From the point of view of the programmer of an application, Chapter 5 introduced two quite different kinds of concurrency coordination requirements: *sequence coordination* and *before-or-after atomicity*. Sequence coordination is a constraint of the type "Action *W* must happen before action *X*". For correctness, the first action must complete before the second action begins. For example, reading of typed characters from a keyboard must happen before running the program that presents those characters on a display. As a general rule, when writing a program one can anticipate the sequence coordination constraints, and the programmer knows the identity of the concurrent actions. Sequence coordination thus is usually explicitly programmed, using either special language constructs or shared variables such as the eventcounts of Chapter 5.

In contrast, *before-or-after atomicity* is a more general constraint that several actions that concurrently operate on the same data should not interfere with one another. We define before-or-after atomicity as follows:

Before-or-after atomicity

Concurrent actions have the *before-or-after* property if their effect from the point of view of their invokers is the same as if the actions occurred either *completely before* or *completely after* one another.

In Chapter 5 we saw how before-or-after actions can be created with explicit locks and a thread manager that implements the procedures ACQUIRE and RELEASE. Chapter 5 showed some examples of before-or-after actions using locks, and emphasized that programming correct before-or-after actions, for example coordinating a bounded buffer with several producers or several consumers, can be a tricky proposition. To be confident of correctness, one needs to establish a compelling argument that every action that touches a shared variable follows the locking protocol.

从服务器故障的角度来看，一个自然的模型是将服务器提供的每个动作都设为全有或全无的。

9.1.5 前后原子性：协调并发线程

在第5章中，我们学习了如何通过创建线程来表达并发机会，并发的目标是通过同时运行多个任务来提高性能。此外，上文第9.1.2节指出，中断同样能创造并发。并发线程在其执行路径交叉前不会引发任何特殊问题。路径交叉的方式总能用共享可写数据来描述：并发线程恰好在相近的时间点对同一块可写数据产生兴趣。这些并发线程甚至无需同时运行；若一个线程在操作过程中被挂起（可能由于中断），此时另一个运行中的线程就可能对挂起线程正在（或将来会再次）操作的数据产生兴趣。

从应用程序程序员的角度来看，第五章介绍了两种截然不同的并发协调需求：*sequence coordination* 和 *before-or-after atomicity*。序列协调是一种“动作*w*必须在动作*x*之前发生”的约束类型。为了确保正确性，第一个动作必须在第二个动作开始前完成。例如，从键盘读取输入的字符必须在运行将这些字符显示在屏幕上的程序之前完成。一般来说，编写程序时可以预见序列协调约束，且程序员知晓并发动作的身份。因此，序列协调通常通过特殊语言结构或共享变量（如第五章所述的事件计数器）进行显式编程。

相比之下，*before-or-after atomicity* 是一个更为普遍的约束条件，它要求多个同时操作同一数据的动作不应相互干扰。我们将前后原子性定义如下：

前或后原子性

并发操作具有***before-or-after***属性，如果从调用者的角度来看，其效果与这些操作以***completely before*** 或***completely after***的顺序发生相同。

在第5章中，我们了解了如何通过显式锁和实现ACQUIRE与RELEASE过程的线程管理器来创建先后动作。第5章展示了一些使用锁实现先后动作的实例，并强调编程实现正确的先后动作（例如协调多个生产者或多个消费者的有界缓冲区）可能是个棘手的命题。为确保正确性，必须构建一个令人信服的论证，证明每个涉及共享变量的操作都遵循了锁定协议。

One thing that makes before-or-after atomicity different from sequence coordination is that the programmer of an action that must have the before-or-after property does not necessarily know the identities of all the other actions that might touch the shared variable. This lack of knowledge can make it problematic to coordinate actions by explicit program steps. Instead, what the programmer needs is an automatic, implicit mechanism that ensures proper handling of every shared variable. This chapter will describe several such mechanisms. Put another way, correct coordination requires discipline in the way concurrent threads read and write shared data.

Applications for before-or-after atomicity in a computer system abound. In an operating system, several concurrent threads may decide to use a shared printer at about the same time. It would not be useful for printed lines of different threads to be interleaved in the printed output. Moreover, it doesn't really matter which thread gets to use the printer first; the primary consideration is that one use of the printer be complete before the next begins, so the requirement is to give each print job the before-or-after atomicity property.

For a more detailed example, let us return to the banking application and the TRANSFER procedure. This time the account balances are held in shared memory variables (recall that the declaration keyword **reference** means that the argument is call-by-reference, so that TRANSFER can change the values of those arguments):

```
procedure TRANSFER (reference debit_account, reference credit_account, amount)
    debit_account ← debit_account - amount
    credit_account ← credit_account + amount
```

Despite their unitary appearance, a program statement such as "$X \leftarrow X + Y$" is actually composite: it involves reading the values of X and Y, performing an addition, and then writing the result back into X. If a concurrent thread reads and changes the value of X between the read and the write done by this statement, that other thread may be surprised when this statement overwrites its change.

Suppose this procedure is applied to accounts A (initially containing $300) and B (initially containing $100) as in

```
TRANSFER (A, B, $10)
```

We expect account A, the debit account, to end up with $290, and account B, the credit account, to end up with $110. Suppose, however, a second, concurrent thread is executing the statement

```
TRANSFER (B, C, $25)
```

where account C starts with $175. When both threads complete their transfers, we expect B to end up with $85 and C with $200. Further, this expectation should be fulfilled no matter which of the two transfers happens first. But the variable *credit_account* in the first thread is bound to the same object (account B) as the variable *debit_account* in the second thread. The risk to correctness occurs if the two transfers happen at about the same time. To understand this risk, consider Figure 9.2, which illustrates several possible time sequences of the READ and WRITE steps of the two threads with respect to variable B.

前后原子性与顺序协调的一个不同之处在于，对于必须具有前后原子性属性的操作，其程序员未必知晓所有可能触及共享变量{v*}的其他操作的身份。这种认知缺失使得通过显式程序步骤来协调操作变得棘手。相反，程序员需要的是一种自动、隐式的机制，以确保每个共享变量{v*}都能得到妥善处理。本章将描述几种这样的机制。换言之，正确的协调要求并发线程在读写共享数据时遵循一定的规范。

计算机系统中前后原子性的应用比比皆是。在操作系统中，多个并发线程可能几乎同时决定使用共享打印机。若不同线程的打印行在输出中交错出现，将毫无意义。此外，哪个线程优先使用打印机其实并不重要；关键在于每次打印任务必须完整执行完毕后，下一个任务才能开始。因此，核心需求是为每个打印作业赋予前后原子性属性。

为了更详细的示例，让我们回到银行应用程序和TRANSFER过程。这次账户余额存储在共享内存变量中（记得声明关键字**reference**表示参数是按引用传递的，因此TRANSFER可以改变这些参数的值）：

```
过程 TRANSFER (引用 debit_account, 引用 credit_account, amount)
debit_account ← debit_account - amount  credit_account ← credit_account + amount
```

尽管表面上看似单一，但像“$x \leftarrow x + y$”这样的程序语句实际上是复合的：它涉及读取x和y的值，执行加法运算，然后将结果写回x。如果在此语句的读取和写入操作之间，有并发线程读取并更改了x的值，那么当此语句覆盖其更改时，其他线程可能会感到意外。

假设这一程序应用于账户A（初始金额为300美元）和B（初始金额为100美元），如

```
转账 (A, B, $10)
```

我们预计借方账户A最终会有290美元，而贷方账户B最终会有110美元。然而，假设有第二个并发线程正在执行该语句

```
转账 (B, C, $25)
```

其中账户C初始金额为175美元。当两个线程都完成转账后，我们预期B最终应有85美元，而C应有200美元。此外，无论两个转账操作中哪一个先发生，这一预期都应得到满足。但第一个线程中的变量*credit_account*与第二个线程中的变量*debit_account*绑定到了同一个对象（即账户B）。若两个转账操作几乎同时发生，就会产生正确性风险。要理解这一风险，请参考图9.2，该图展示了两个线程针对变量B的READ和WRITE步骤可能出现的几种时间序列。

With each time sequence the figure shows the history of values of the cell containing the balance of account *B*. If both steps 1–1 and 1–2 precede both steps 2–1 and 2–2, (or vice-versa) the two transfers will work as anticipated, and *B* ends up with $85. If, however, step 2–1 occurs after step 1–1, but before step 1–2, a mistake will occur: one of the two transfers will not affect account *B*, even though it should have. The first two cases illustrate histories of shared variable *B* in which the answers are the correct result; the remaining four cases illustrate four different sequences that lead to two incorrect values for *B*.

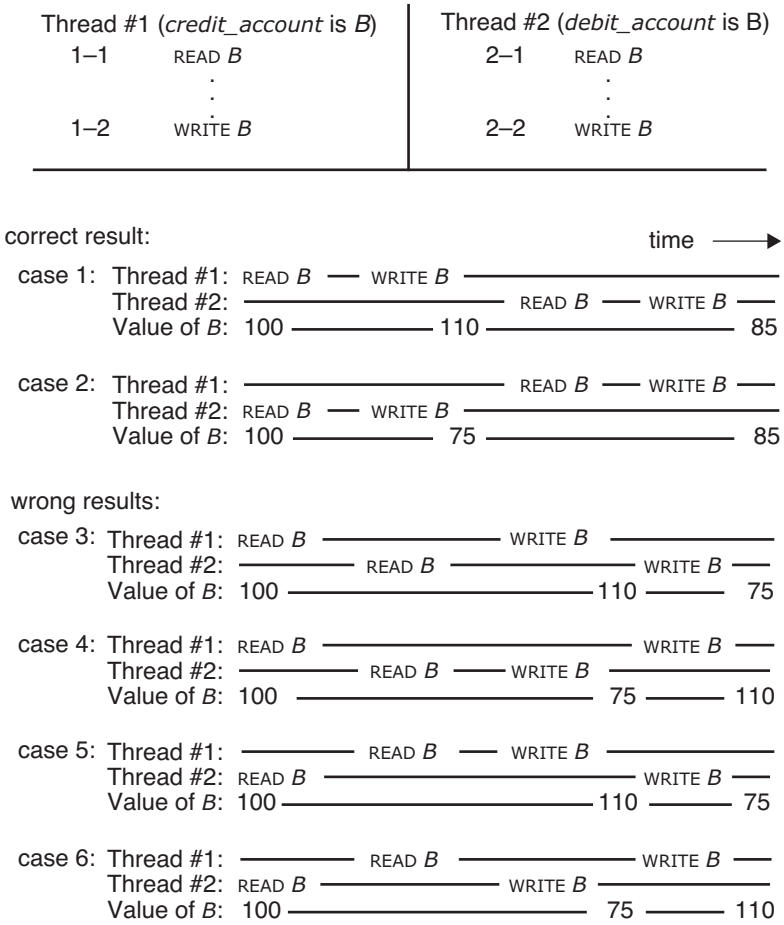

FIGURE 9.2

Six possible histories of variable *B* if two threads that share *B* do not coordinate their concurrent activities.

图中每个时间序列展示了包含账户余额*B*的单元格数值历史。若步骤1-1和1-2均先于步骤2-1和2-2执行（或反之），两次转账将按预期运作，最终*B*余额为85美元。然而，若步骤2-1在步骤1-1之后、步骤1-2之前执行，则会出现错误：其中一笔转账本应作用于账户*B*却未能生效。前两种情况展示了共享变量*B*产生正确结果的历史记录；其余四种情况则呈现了导致*B*出现两种错误值的不同执行序列。

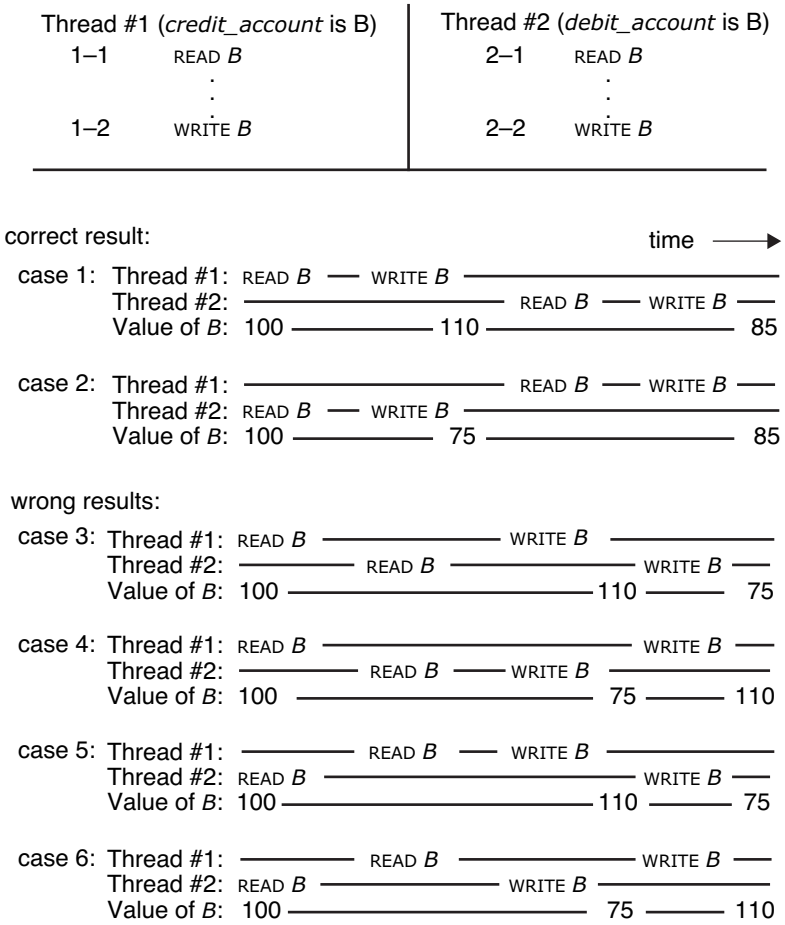

图9.2

变量B的六种可能历史，如果共享B的两个线程不协调它们的并发活动。

Thus our goal is to ensure that one of the first two time sequences actually occurs. One way to achieve this goal is that the two steps 1–1 and 1–2 should be atomic, and the two steps 2–1 and 2–2 should similarly be atomic. In the original program, the steps

debit_account ← *debit_account - amount*

and

credit_account ← *credit_account + amount*

should each be atomic. There should be no possibility that a concurrent thread that intends to change the value of the shared variable *debit_account* read its value between the READ and WRITE steps of this statement.

9.1.6 Correctness and Serialization

The notion that the first two sequences of Figure 9.2 are correct and the other four are wrong is based on our understanding of the banking application. It would be better to have a more general concept of correctness that is independent of the application. Application independence is a modularity goal: we want to be able to make an argument for correctness of the mechanism that provides before-or-after atomicity without getting into the question of whether or not the application using the mechanism is correct.

There is such a correctness concept: coordination among concurrent actions can be considered to be correct *if every result is guaranteed to be one that could have been obtained by some purely serial application* of those same actions.

The reasoning behind this concept of correctness involves several steps. Consider Figure 9.3,which shows, abstractly, the effect of applying some action, whether atomic or not, to a system: the action changes the state of the system. Now, if we are sure that:

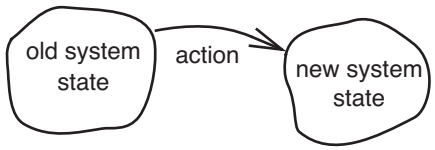

FIGURE 9.3

A single action takes a system from one state to another state.

1. the old state of the system was correct from the point of view of the application, and
2. the action, performing all by itself, correctly transforms any correct old state to a correct new state,

then we can reason that the new state must also be correct. This line of reasoning holds for any application-dependent definition of "correct" and "correctly transform", so our reasoning method is independent of those definitions and thus of the application.

The corresponding requirement when several actions act concurrently, as in Figure 9.4, is that the resulting new state ought to be one of those that would have resulted from some serialization of the several actions, as in Figure 9.5. This correctness criterion means that concurrent actions are correctly coordinated if their result is guaranteed to be one that would have been obtained by *some* purely serial application of those same actions.

因此，我们的目标是确保前两个时间序列之一实际发生。实现这一目标的一种方法是，步骤1-1和1-2应当是原子性的，同样地，步骤2-1和2-2也应当是原子性的。在原程序中，这些步骤

debit_account ← *debit_account- amount*

和

credit_account ← *credit_account+ amount*

每个操作都应是原子性的。必须确保不存在这样的可能性：一个意图更改共享变量*debit_account*值的并发线程，在本语句的READ和WRITE步骤之间读取到该变量的值。

9.1.6 正确性与序列化

图9.2中前两个序列正确而其余四个错误的观点，是基于我们对银行应用的理解。若能建立一个不依赖于具体应用的、更普适的正确性概念则更为理想。应用独立性是模块化的目标：我们希望能在不涉及使用该机制的应用是否正确的前提下，单独论证提供"前或后"原子性的机制本身的正确性。

存在这样一种正确性概念：并发动作间的协调可被视为这些动作本身的正确性*if every result is guaranteed to be one that could have been obtained by some purely serial application*。

这一正确性概念背后的推理涉及几个步骤。考虑图9.3，它抽象地展示了这种影响

对系统施加某个动作（无论是原子性的还是非原子性的）时：该动作会改变系统的状态。现在，如果我们确信：

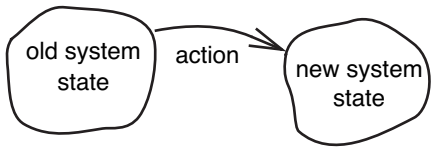

图9.3

单个动作将系统从一个状态转移到另一个状态。

1. 系统的旧状态是正确的
   从应用的角度来看，

2. 这一行动，完全由自身执行，
   正确地将任何正确的旧状态转换为正确的新状态，

那么我们可以推断新状态也必定是正确的。这一推理思路适用于任何依赖于应用的“正确”和“正确转换”定义，因此我们的推理方法与这些定义无关，从而与应用本身无关。

当多个动作并发执行时，如图9.4所示，其对应的要求是：最终的新状态应当是这些动作按某种顺序串行执行时可能产生的状态之一，如图9.5所示。这一正确性标准意味着，若并发动作的结果能确保等同于这些相同动作*some*纯粹串行应用所获得的结果，则说明并发动作的协调是正确的。

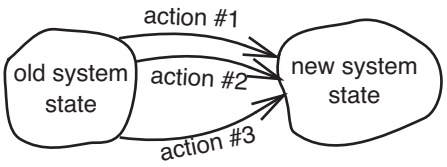

FIGURE 9.4

When several actions act concurrently, they together produce a new state. If the actions are before-or-after and the old state was correct, the new state will be correct.

So long as the only coordination requirement is before-or-after atomicity, any serialization will do.

Moreover, we do not even need to insist that the system actually traverse the intermediate states along any particular path of Figure 9.5—it may instead follow the dotted trajectory through intermediate states that are not by themselves correct, according to the application's definition. As long as the intermediate states are not visible above the implementing layer, and the system is guaranteed to end up in one of the acceptable final states, we can declare the coordination to be correct because there exists a trajectory that leads to that state for which a correctness argument could have been applied to every step.

Since our definition of before-or-after atomicity is that each before-or-after action act as though it ran either completely before or completely after each other before-or-after action, before-or-after atomicity leads directly to this concept of correctness. Put another way, before-or-after atomicity has the effect of serializing the actions, so it follows that before-or-after atomicity guarantees correctness of coordination. A different way of

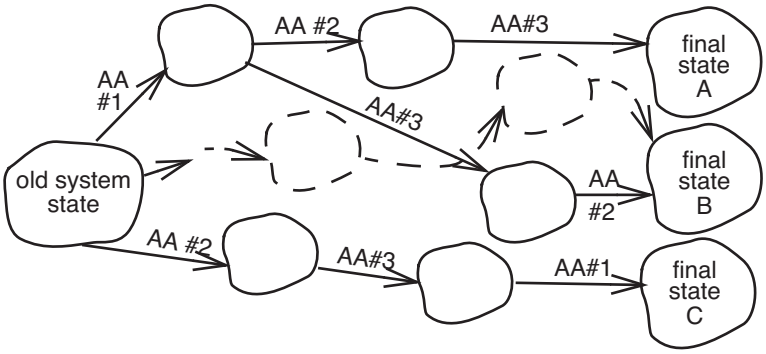

FIGURE 9.5

We insist that the final state be one that could have been reached by some serialization of the atomic actions, but we don't care which serialization. In addition, we do not need to insist that the intermediate states ever actually exist. The actual state trajectory could be that shown by the dotted lines, *but only if there is no way of observing the intermediate states from the outside*.

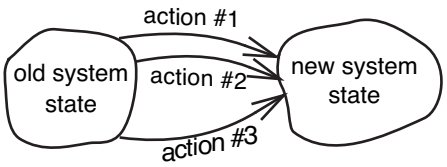

图9.4

当多个操作并发执行时，它们共同产生一个新状态。若这些操作满足前后顺序且旧状态正确，则新状态也将保持正确。

只要协调需求仅限于前后原子性，任何序列化{v*}方法都可行。

此外，我们甚至无需坚持系统必须实际遍历图9.5中任何特定路径上的中间状态——根据应用的定义，系统可以转而遵循虚线轨迹通过那些本身并不正确的中间状态。只要这些中间状态在实现层之上不可见，且系统最终必定会达到某个可接受的最终状态，我们就可以宣称该协调是正确的，因为存在一条通往该状态的轨迹，且该轨迹的每一步本都可以应用正确性论证。

由于我们对“前后原子性”的定义是，每一个前后动作都表现得像是在其他所有前后动作之前或之后完全执行，因此前后原子性直接引出了这一正确性概念。换言之，前后原子性的作用在于将动作序列化，由此可推知，前后原子性保证了协调的正确性。另一种表述方式是

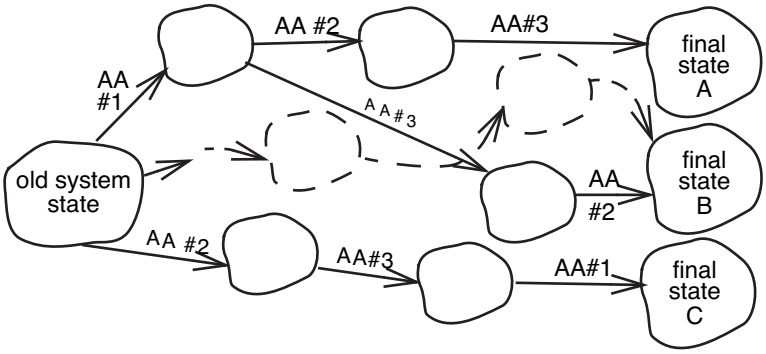

图9.5

我们坚持认为，最终状态必须能够通过某种原子操作的串行化序列达成，但具体是哪一种串行化序列则无关紧要。此外，我们无需强求中间状态必须真实存在。实际的状态轨迹可能如虚线所示，但前提是从外部无法观测到这些中间状态。

expressing this idea is to say that when concurrent actions have the before-or-after property, they are *serializable: there exists some serial order of those concurrent transactions that would, if followed, lead to the same ending state.*[*] Thus in Figure 9.2, the sequences of case 1 and case 2 could result from a serialized order, but the actions of cases 3 through 6 could not.

In the example of Figure 9.2, there were only two concurrent actions and each of the concurrent actions had only two steps. As the number of concurrent actions and the number of steps in each action grows there will be a rapidly growing number of possible orders in which the individual steps can occur, but only some of those orders will ensure a correct result. Since the purpose of concurrency is to gain performance, one would like to have a way of choosing from the set of correct orders the one correct order that has the highest performance. As one might guess, making that choice can in general be quite difficult. In Sections 9.4 and 9.5 of this chapter we will encounter several programming disciplines that ensure choice from a subset of the possible orders, all members of which are guaranteed to be correct but, unfortunately, may not include the correct order that has the highest performance.

In some applications it is appropriate to use a correctness requirement that is stronger than serializability. For example, the designer of a banking system may want to avoid anachronisms by requiring what might be called *external time consistency*: if there is any external evidence (such as a printed receipt) that before-or-after action T_1 ended before before-or-after action T_2 began, the serialization order of T_1 and T_2 inside the system should be that T_1 precedes T_2. For another example of a stronger correctness requirement, a processor architect may require *sequential consistency*: when the processor concurrently performs multiple instructions from the same instruction stream, the result should be as if the instructions were executed in the original order specified by the programmer.

Returning to our example, a real funds-transfer application typically has several distinct before-or-after atomicity requirements. Consider the following auditing procedure; its purpose is to verify that the sum of the balances of all accounts is zero (in double-entry bookkeeping, accounts belonging to the bank, such as the amount of cash in the vault, have negative balances):

```
procedure AUDIT()
    sum ← 0
    for each W ← in bank.accounts
        sum ← sum + W.balance
    if (sum ≠ 0) call for investigation
```

Suppose that AUDIT is running in one thread at the same time that another thread is transferring money from account *A* to account *B*. If AUDIT examines account *A* before the transfer and account *B* after the transfer, it will count the transferred amount twice and

[*] The general question of whether or not a collection of existing transactions is serializable is an advanced topic that is addressed in database management. Problem set *36* explores one method of answering this question.

表达这一观点的一种方式是，当并发操作具有先后关系属性时，它们是 *serializable: there exists some serial order of those concurrent transactions that would, if followed, lead to the same ending state.**。因此，在图9.2中，情况1和情况2的序列可能源自串行化顺序，但情况3至6的操作则不可能。

在图9.2的示例中，仅存在两个并发动作，且每个并发动作仅包含两个步骤。随着并发动作数量及每个动作中步骤数的增加，各步骤可能发生的顺序组合将呈指数级增长，但其中仅有部分顺序能确保结果正确。由于并发的本质在于提升性能，我们期望能从所有正确顺序中筛选出性能最优的那个正确顺序。可以想见，这一选择过程通常极具挑战性。本章9.4与9.5节将介绍若干编程规范，这些规范确保从可能的顺序子集中进行选择——该子集所有成员均能保证正确性，但遗憾的是，可能不包含性能最优的正确顺序。

在某些应用中，采用比可串行化更强的正确性要求是合适的。例如，银行系统的设计者可能希望通过要求所谓的*external time consistency*来避免时代错位：若存在任何外部证据（如打印的收据）表明"之前或之后动作"T_1在"之前或之后动作"T_2开始前已结束，则系统内部T_1与T_2的串行化顺序应确保T_1先于T_2。另一个更强正确性要求的例子是，处理器架构师可能要求 *sequential consistency*：当处理器并发执行同一指令流中的多条指令时，其结果应如同这些指令按照程序员指定的原始顺序依次执行。

回到我们的例子，一个真实的资金转账应用通常有几个明确的前后原子性要求。考虑以下审计程序；其目的是验证所有账户余额之和为零（在复式记账法中，属于银行的账户，如金库中的现金金额，具有负余额）：

```
                程序 审计() sum ← 0 对每一个 W ← 在
                bank.accounts + W.balance 如果 (
sum ← sum       sum ≠ 0) 则要求调查
```

假设AUDIT在一个线程中运行的同时，另一个线程正在将资金从账户A转移到账户B。如果AUDIT在转账前检查账户A，在转账后检查账户B，就会将转账金额重复计算两次，

* The general question of whether or not a collection of existing transactions is serializable is an advanced topic that is addressed in database management. Problem set *36* explores one method of answering this question.

thus will compute an incorrect answer. So the entire auditing procedure should occur either before or after any individual transfer: we want it to be a before-or-after action.

There is yet another before-or-after atomicity requirement: if AUDIT should run after the statement in TRANSFER

debit_account ← *debit_account* - *amount*

but before the statement

credit_account ← *credit_account* + *amount*

it will calculate a sum that does not include *amount*; we therefore conclude that the two balance updates should occur either completely before or completely after any AUDIT action; put another way, TRANSFER should be a before-or-after action.

9.1.7 All-or-Nothing and Before-or-After Atomicity

We now have seen examples of two forms of atomicity: all-or-nothing and before-or-after. These two forms have a common underlying goal: to hide the internal structure of an action. With that insight, it becomes apparent that atomicity is really a unifying concept:

Atomicity

An action is atomic if there is no way for a higher layer to discover the internal structure of its implementation.

This description is really the fundamental definition of atomicity. From it, one can immediately draw two important consequences, corresponding to all-or-nothing atomicity and to before-or-after atomicity:

1. From the point of view of a procedure that invokes an atomic action, the atomic action always appears either to complete as anticipated, or to do nothing. This consequence is the one that makes atomic actions useful in recovering from failures.
2. From the point of view of a concurrent thread, an atomic action acts as though it occurs either *completely before* or *completely after* every other concurrent atomic action. This consequence is the one that makes atomic actions useful for coordinating concurrent threads.

These two consequences are not fundamentally different. They are simply two perspectives, the first from other modules within the thread that invokes the action, the second from other threads. Both points of view follow from the single idea that the internal structure of the action is not visible outside of the module that implements the action. Such hiding of internal structure is the essence of modularity, but atomicity is an exceptionally strong form of modularity. Atomicity hides not just the details of which

因此将计算出错误的答案。因此，整个审计程序应在任何单次转账之前或之后进行：我们希望它是一个“前或后”的动作。

还有另一个前后原子性要求：如果AUDIT应在TRANSFER 中的语句之后运行

debit_account ← *debit_account- amount*

但在声明之前

credit_account ← *credit_account+ amount*

它将计算一个不包含*amount*的总和；因此我们得出结论，这两个余额更新应该完全在任何AUDIT 操作之前或之后发生；换句话说，TRANSFER应该是一个“之前或之后”的操作。

9.1.7 全有或全无与前后原子性

我们现在已经看到了原子性的两种形式示例：全有或全无（all-or-nothing）与前或后（before-or-after）。这两种形式有一个共同的根本目标：隐藏动作的内部结构。有了这一洞见，原子性显然是一个统一的概念：

原子性

An action is atomic if there is no way for a higher layer to discover the internal structure of its implementation.

这一描述实质上是对原子性的根本定义。由此，我们可以立即得出两个重要推论，分别对应全有或全无的原子性以及前后一致的原子性：

1. 从调用原子操作的过程来看，原子操作总是要么如期完成，要么什么都不做。这一特性使得原子操作在从故障中恢复时非常有用。 2. 从并发线程的角度看，原子操作的行为表现为它要么发生在所有其他并发原子操作之前*completely before*，要么之后*completely after*。这一特性使得原子操作在协调并发线程时非常有用。

这两点后果并无本质区别。它们只是两种视角：第一种来自调用该操作的线程内部其他模块，第二种则来自其他线程。两种观点都源于同一个核心理念——操作的内部结构对实现该操作的模块之外不可见。这种内部结构的隐藏正是模块化的精髓，但原子性是一种异常强大的模块化形式。原子性不仅隐藏了具体哪些{v*}

steps form the atomic action, but the very fact that it has structure. There is a kinship between atomicity and other system-building techniques such as data abstraction and client/server organization. Data abstraction has the goal of hiding the internal structure of data; client/server organization has the goal of hiding the internal structure of major subsystems. Similarly, atomicity has the goal of hiding the internal structure of an action. All three are methods of enforcing industrial-strength modularity, and thereby of guaranteeing absence of unanticipated interactions among components of a complex system.

We have used phrases such as "from the point of view of the invoker" several times, suggesting that there may be another point of view from which internal structure *is* apparent. That other point of view is seen by the implementer of an atomic action, who is often painfully aware that an action is actually composite, and who must do extra work to hide this reality from the higher layer and from concurrent threads. Thus the interfaces between layers are an essential part of the definition of an atomic action, and they provide an opportunity for the implementation of an action to operate in any way that ends up providing atomicity.

There is one more aspect of hiding the internal structure of atomic actions: atomic actions can have benevolent side effects. A common example is an audit log, where atomic actions that run into trouble record the nature of the detected failure and the recovery sequence for later analysis. One might think that when a failure leads to backing out, the audit log should be rolled back, too; but rolling it back would defeat its purpose—the whole point of an audit log is to record details about the failure. The important point is that the audit log is normally a private record of the layer that implemented the atomic action; in the normal course of operation it is not visible above that layer, so there is no requirement to roll it back. (A separate atomicity requirement is to ensure that the log entry that describes a failure is complete and not lost in the ensuing recovery.)

Another example of a benevolent side effect is performance optimization. For example, in a high-performance data management system, when an upper layer atomic action asks the data management system to insert a new record into a file, the data management system may decide as a performance optimization that now is the time to rearrange the file into a better physical order. If the atomic action fails and aborts, it need ensure only that the newly-inserted record be removed; the file does not need to be restored to its older, less efficient, storage arrangement. Similarly, a lower-layer cache that now contains a variable touched by the atomic action does not need to be cleared and a garbage collection of heap storage does not need to be undone. Such side effects are not a problem, as long as they are hidden from the higher-layer client of the atomic action except perhaps in the speed with which later actions are carried out, or across an interface that is intended to report performance measures or failures.

步骤构成了原子动作，但关键在于它本身具有结构。原子性与其它系统构建技术，如数据抽象和客户端/服务器组织，存在亲缘关系。数据抽象旨在隐藏数据的内部结构；客户端/服务器组织旨在隐藏主要子系统的内部结构。同样地，原子性旨在隐藏动作的内部结构。这三者都是实施工业级模块化的方法，从而确保复杂系统各组件间不会出现意料之外的交互。

我们多次使用了“从调用者的角度来看”这样的表述，暗示可能存在另一种视角，从中可以观察到内部结构*is*。这一视角由原子动作的实现者所持有，他们往往痛苦地意识到一个动作实际上是复合的，并且必须额外努力向更高层级及并发线程隐藏这一事实。因此，层级间的接口是定义原子动作不可或缺的部分，它们为动作的实现提供了以任何方式运作的可能性，只要最终能保证原子性。

隐藏原子动作内部结构还有另一个方面：原子动作可以具有良性的副作用。一个常见的例子是审计日志，其中遇到问题的原子动作会记录检测到的故障性质及恢复序列，以供后续分析。有人可能认为，当故障导致回滚时，审计日志也应一并回滚；但这样做将违背其初衷——审计日志的全部意义就在于记录故障的细节。关键在于，审计日志通常是实现原子动作的层的私有记录；在正常操作过程中，它对上层不可见，因此无需回滚。（另一项原子性要求是确保描述故障的日志条目完整无缺，不会在随后的恢复过程中丢失。）

另一个良性副作用的例子是性能优化。例如，在一个高性能数据管理系统中，当上层原子操作要求系统向文件插入新记录时，数据管理系统可能出于性能优化考虑，决定此时对文件进行物理重排以获得更优存储顺序。若原子操作失败并中止，只需确保新插入的记录被移除即可，无需将文件恢复至原先效率较低的存储布局。同样地，底层缓存若已载入被原子操作访问的变量，则无需清空缓存；堆存储的垃圾回收操作也无需回退。只要这些副作用对原子操作的高层客户端透明——除了可能影响后续操作的执行速度，或是通过专门报告性能指标或故障的接口暴露——它们就不会构成问题。

9.2 All-or-Nothing Atomicity I: Concepts

Section 9.1 of this chapter defined the goals of all-or-nothing atomicity and before-or-after atomicity, and provided a conceptual framework that at least in principle allows a designer to decide whether or not some proposed algorithm correctly coordinates concurrent activities. However, it did not provide any examples of actual implementations of either goal. This section of the chapter, together with the next one, describe some widely applicable techniques of systematically implementing *all-or-nothing* atomicity. Later sections of the chapter will do the same for before-or-after atomicity.

Many of the examples employ the technique introduced in Chapter 5 called *bootstrapping*, a method that resembles inductive proof. To review, bootstrapping means to first look for a systematic way to reduce a general problem to some much-narrowed particular version of that same problem. Then, solve the narrow problem using some specialized method that might work only for that case because it takes advantage of the specific situation. The general solution then consists of two parts: a special-case technique plus a method that systematically reduces the general problem to the special case. Recall that Chapter 5 tackled the general problem of creating before-or-after actions from arbitrary sequences of code by implementing a procedure named ACQUIRE that itself required before-or-after atomicity of two or three lines of code where it reads and then sets a lock value. It then implemented that before-or-after action with the help of a special hardware feature that directly makes a before-or-after action of the read and set sequence, and it also exhibited a software implementation (in Sidebar 5.2) that relies only on the hardware performing ordinary LOADs and STOREs as before-or-after actions. This chapter uses bootstrapping several times. The first example starts with the special case and then introduces a way to reduce the general problem to that special case. The reduction method, called the *version history*, is used only occasionally in practice, but once understood it becomes easy to see why the more widely used reduction methods that will be described in Section 9.3 work.

9.2.1 Achieving All-or-Nothing Atomicity: ALL_OR_NOTHING_PUT

The first example is of a scheme that does an all-or-nothing update of a single disk sector. The problem to be solved is that if a system crashes in the middle of a disk write (for example, the operating system encounters a bug or the power fails), the sector that was being written at the instant of the failure may contain an unusable muddle of old and new data. The goal is to create an all-or-nothing PUT with the property that when GET later reads the sector, it always returns either the old or the new data, but never a muddled mixture.

To make the implementation precise, we develop a disk fault tolerance model that is a slight variation of the one introduced in Chapter 8[on-line], taking as an example application a calendar management program for a personal computer. The user is hoping that, if the system fails while adding a new event to the calendar, when the system later restarts the calendar will be safely intact. Whether or not the new event ended up in the

9.2 全有或全无原子性 I：概念

本章第9.1节定义了全有或全无原子性及前后原子性的目标，并提供了一个概念框架，至少在原则上允许设计者判断某个提议的算法是否正确协调了并发活动。然而，它并未提供实现任一目标的具体实例。本章的这一节连同下一节，将描述一些广泛适用的系统化实现*all-or-nothing*原子性的技术。本章后续章节将对前后原子性进行同样的探讨。

许多示例采用了第5章介绍的名为*boot-strapping*的技术，这种方法类似于归纳证明。回顾一下，自举意味着首先寻找一种系统化的方法，将一般性问题简化为同一问题的某个更为狭窄的特定版本。然后，利用某种专门的方法解决这个狭窄问题，该方法可能仅适用于该特定情况，因为它利用了具体情境的优势。通用解决方案因此由两部分组成：一个针对特殊情形的技术，加上一个将一般问题系统化归约为特殊情形的方法。 回顾5.2.3节通过实现一个名为ACQUIRE的过程，解决了从任意代码序列创建前后动作的通用问题。该过程本身要求在读取并设置锁值的两三行代码中具备前后原子性。随后，它借助一个特殊的硬件特性直接实现了该读写序列的前后动作，并展示了仅依赖硬件执行常规LOAD和STORE作为前后动作的软件实现（见边栏5.2）。 本章多次运用自举方法。第一个示例从特殊情形出发，随后引入将一般问题归约至该特殊情形的方法。这种被称为*version history*的归约方法在实践中使用频率不高，但一旦理解其原理，就能轻松领会第9.3节将介绍的更广泛使用的归约方法为何有效。

9.2.1 实现全有或全无原子性：ALL_OR_NOTHING_PUT

第一个例子是关于一种对单个磁盘扇区进行全有或全无更新的方案。要解决的问题是，如果系统在磁盘写入过程中崩溃（例如，操作系统遇到错误或电源故障），那么在故障瞬间正在写入的扇区可能会包含新旧数据的混乱混合，导致无法使用。目标是创建一个具有全有或全无特性的PUT，使得当GET后续读取该扇区时，总是返回旧数据或新数据，而绝不会是混乱的混合体。

为了使实现更加精确，我们开发了一个磁盘容错模型，该模型是对第8章[在线]中介绍的模型稍作变动的版本，并以个人电脑上的日历管理程序作为示例应用。用户期望的是，如果在向日历添加新事件时系统发生故障，当系统稍后重启时，日历仍能保持完好无损。无论新事件是否最终成功添加到日历中，

calendar is less important than that the calendar not be damaged by inopportune timing of the system failure. This system comprises a human user, a display, a processor, some volatile memory, a magnetic disk, an operating system, and the calendar manager program. We model this system in several parts:

Overall system fault tolerance model.

- error-free operation: All work goes according to expectations. The user initiates actions such as adding events to the calendar and the system confirms the actions by displaying messages to the user.
- tolerated error: The user who has initiated an action notices that the system failed before it confirmed completion of the action and, when the system is operating again, checks to see whether or not it actually performed that action.
- untolerated error: The system fails without the user noticing, so the user does not realize that he or she should check or retry an action that the system may not have completed.

The tolerated error specification means that, to the extent possible, the entire system is fail-fast: if something goes wrong during an update, the system stops before taking any more requests, and the user realizes that the system has stopped. One would ordinarily design a system such as this one to minimize the chance of the untolerated error, for example by requiring supervision by a human user. The human user then is in a position to realize (perhaps from lack of response) that something has gone wrong. After the system restarts, the user knows to inquire whether or not the action completed. This design strategy should be familiar from our study of best effort networks in Chapter 7[on-line]. The lower layer (the computer system) is providing a best effort implementation. A higher layer (the human user) supervises and, when necessary, retries. For example, suppose that the human user adds an appointment to the calendar but just as he or she clicks "save" the system crashes. The user doesn't know whether or not the addition actually succeeded, so when the system comes up again the first thing to do is open up the calendar to find out what happened.

Processor, memory, and operating system fault tolerance model.

This part of the model just specifies more precisely the intended fail-fast properties of the hardware and operating system:

- error-free operation: The processor, memory, and operating system all follow their specifications.
- detected error: Something fails in the hardware or operating system. The system is fail-fast: the hardware or operating system detects the failure and restarts from a clean slate *before* initiating any further PUTs to the disk.
- untolerated error: Something fails in the hardware or operating system. The processor muddles along and PUTs corrupted data to the disk before detecting the failure.

日历的重要性不及系统故障的不合时宜对日历造成的损害。该系统由人类用户、显示器、处理器、一些易失性内存、磁盘、操作系统以及日历管理程序组成。我们将该系统建模为几个部分：

Overall system fault tolerance model.

- 无差错操作：所有工作均按预期进行。用户发起诸如向日历添加事件等操作，系统通过向用户显示消息来确认这些操作。
- 容忍误差：发起操作的用户注意到系统在确认操作完成之前已发生故障，当系统再次运行时，会检查该操作是否实际执行。
- 不可容忍的错误：系统在用户未察觉的情况下发生故障，导致用户意识不到应该检查或重试某个可能未被系统完成的动作。

容错规范意味着，在可能的情况下，整个系统应实现快速失效：如果在更新过程中出现问题，系统会在接收更多请求之前停止运行，用户也会意识到系统已停止。通常，人们会设计这样的系统以尽量减少不可容忍错误的发生概率，例如通过要求人工用户的监督。这样，人工用户便能察觉（可能是由于缺乏响应）出问题了。系统重启后，用户知道去查询操作是否完成。这一设计策略应与我们第7章[在线]中对尽力而为网络的研究相呼应。底层（计算机系统）提供尽力而为的实现，而高层（人工用户）进行监督并在必要时重试。例如，假设人工用户向日历添加一个约会，就在点击“保存”时系统崩溃。用户不知道添加是否真的成功，因此系统再次启动后，第一件事就是打开日历查看发生了什么。

Processor, memory, and operating system fault tolerance model.

模型的这一部分只是更精确地规定了硬件和操作系统预期的快速失效特性：

- 无差错运行：处理器、内存和操作系统均遵循其规格要求。
- 检测到错误：硬件或操作系统出现故障。系统采用快速失效机制：硬件或操作系统检测到故障后，会从干净状态*before*重启，避免对磁盘发起任何进一步的PUT操作。
- 不可容忍的错误：硬件或操作系统中出现了故障。处理器在检测到故障前继续混乱运行，并将PUT损坏的数据写入磁盘。

The primary goal of the processor/memory/operating-system part of the model is to detect failures and stop running before any corrupted data is written to the disk storage system. The importance of detecting failure before the next disk write lies in error containment: if the goal is met, the designer can assume that the only values potentially in error must be in processor registers and volatile memory, and the data on the disk should be safe, with the exception described in Section 8.5.4.2: if there was a PUT to the disk in progress at the time of the crash, the failing system may have corrupted the disk buffer in volatile memory, and consequently corrupted the disk sector that was being written.

The recovery procedure can thus depend on the disk storage system to contain only uncorrupted information, or at most one corrupted disk sector. In fact, after restart the disk will contain the *only* information. "Restarts from a clean slate" means that the system discards all state held in volatile memory. This step brings the system to the same state as if a power failure had occurred, so a single recovery procedure will be able to handle both system crashes and power failures. Discarding volatile memory also means that all currently active threads vanish, so everything that was going on comes to an abrupt halt and will have to be restarted.

Disk storage system fault tolerance model.

Implementing all-or-nothing atomicity involves some steps that resemble the decay masking of MORE_DURABLE_PUT/GET in Chapter 8[on-line]—in particular, the algorithm will write multiple copies of data. To clarify how the all-or-nothing mechanism works, we temporarily back up to CAREFUL_PUT/GET (see Section 8.5.4.5), which masks soft disk errors but not hard disk errors or disk decay. To simplify further, we pretend for the moment that a disk never decays and that it has no hard errors. (Since this perfect-disk assumption is obviously unrealistic, we will reverse it in Section 9.7, which describes an algorithm for all-or-nothing atomicity despite disk decay and hard errors.)

With the perfect-disk assumption, only one thing can go wrong: a system crash at just the wrong time. The fault tolerance model for this simplified careful disk system then becomes:

- error-free operation: CAREFUL_GET returns the result of the most recent call to CAREFUL_PUT at *sector_number* on *track*, with *status* = OK.
- detectable error: The operating system crashes during a CAREFUL_PUT and corrupts the disk buffer in volatile storage, and CAREFUL_PUT writes corrupted data on one sector of the disk.

We can classify the error as "detectable" if we assume that the application has included with the data an end-to-end checksum, calculated before calling CAREFUL_PUT and thus before the system crash could have corrupted the data.

The change in this revision of the careful storage layer is that when a system crash occurs, one sector on the disk may be corrupted, but the client of the interface is confident that (1) that sector is the only one that may be corrupted and (2) if it has been corrupted, any later reader of that sector will detect the problem. Between the processor model and the storage system model, all anticipated failures now lead to the same situa-

该模型中处理器/内存/操作系统部分的主要目标是在任何损坏数据被写入磁盘存储系统之前检测故障并停止运行。关键在于在下一次磁盘写入前捕捉故障以实现错误隔离：若达成此目标，设计者可假定潜在错误值仅存在于处理器寄存器和易失性内存中，而磁盘数据应是安全的——除第8.5.4.2节所述例外情况：若系统崩溃时正在进行PUT写入磁盘的操作，故障系统可能已损毁易失性内存中的磁盘缓冲区，进而破坏了正在写入的磁盘扇区。

因此，恢复过程可以依赖于磁盘存储系统仅包含未损坏的信息，或至多一个损坏的磁盘扇区。实际上，重启后磁盘将包含*only*信息。“从零开始重启”意味着系统丢弃了易失性内存中保存的所有状态。这一步使系统达到与发生电源故障相同的状态，因此单一的恢复程序能够同时处理系统崩溃和电源故障。丢弃易失性内存还意味着所有当前活跃的线程消失，所以正在进行的一切都会突然停止，必须重新启动。

Disk storage system fault tolerance model.

实现全有或全无原子性涉及一些步骤，这些步骤类似于第8章[在线]中MORE_DURABLE_PUT/GET的衰变掩蔽——特别是算法会写入数据的多个副本。为了阐明全有或全无机制的工作原理，我们暂时回溯到CAREFUL_PUT/GET（见8.5.4.5节），它掩盖了软磁盘错误，但不包括硬盘错误或磁盘衰变。为了进一步简化，我们暂时假设磁盘永不衰变且不存在硬错误。（由于这种完美磁盘假设显然不切实际，我们将在第9.7节中予以推翻，该节描述了一种即使面对磁盘衰变和硬错误也能实现全有或全无原子性的算法。）

在完美磁盘假设下，只有一种情况可能出错：系统恰好在最糟糕的时刻崩溃。因此，这种简化谨慎磁盘系统的容错模型就变为：

- 无差错操作：CAREFUL_GET 返回最近一次在 *track* 上通过 *sector_number* 调用 CAREFUL_PUT 的结果，附带 *status* = OK。
- 可检测错误：操作系统在CAREFUL_PUT期间崩溃，损坏了易失性存储中的磁盘缓冲区，且CAREFUL_PUT将损坏的数据写入磁盘的一个扇区。

如果假设应用程序在调用CAREFUL_PUT 之前（即系统崩溃可能导致数据损坏之前）已经计算并包含了端到端的校验和，那么我们可以将此类错误归类为“可检测的”。

本次谨慎存储层修订的变动在于，当系统崩溃发生时，磁盘上的一个扇区可能损坏，但接口客户端确信：(1)该扇区是唯一可能受损的扇区；(2)若该扇区已损坏，任何后续读取该扇区的操作都将检测到问题。在处理器模型与存储系统模型之间，所有预期故障现在都会导致相同的情{v*}

```
1  procedure ALMOST_ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)
2      CAREFUL_PUT (data, all_or_nothing_sector.S1)
3      CAREFUL_PUT (data, all_or_nothing_sector.S2)          // Commit point.
4      CAREFUL_PUT (data, all_or_nothing_sector.S3)

5  procedure ALL_OR_NOTHING_GET (reference data, all_or_nothing_sector)
6      CAREFUL_GET (data1, all_or_nothing_sector.S1)
7      CAREFUL_GET (data2, all_or_nothing_sector.S2)
8      CAREFUL_GET (data3, all_or_nothing_sector.S3)
9      if data1 = data2 then data ← data1                    // Return new value.
10     else data ← data3                                     // Return old value.
```

FIGURE 9.6

Algorithms for ALMOST_ALL_OR_NOTHING_PUT and ALL_OR_NOTHING_GET.

tion: the system detects the failure, resets all processor registers and volatile memory, forgets all active threads, and restarts. No more than one disk sector is corrupted.

Our problem is now reduced to providing the all-or-nothing property: the goal is to create *all-or-nothing disk storage*, which guarantees either to change the data on a sector completely and correctly or else appear to future readers not to have touched it at all. Here is one simple, but somewhat inefficient, scheme that makes use of virtualization: assign, for each data sector that is to have the all-or-nothing property, three physical disk sectors, identified as *S1*, *S2*, and *S3.* The three physical sectors taken together are a virtual "all-or-nothing sector". At each place in the system where this disk sector was previously used, replace it with the all-or-nothing sector, identified by the triple {*S1*, *S2*, *S3*}. We start with an almost correct all-or-nothing implementation named ALMOST_ALL_OR_NOTHING_PUT, find a bug in it, and then fix the bug, finally creating a correct ALL_OR_NOTHING_PUT.

When asked to write data, ALMOST_ALL_OR_NOTHING_PUT writes it three times, on *S1*, *S2*, and *S3,* in that order, each time waiting until the previous write finishes, so that if the system crashes only one of the three sectors will be affected. To read data, ALL_OR_NOTHING_GET reads all three sectors and compares their contents. If the contents of *S1* and *S2* are identical, ALL_OR_NOTHING_GET returns that value as the value of the all-or-nothing sector. If *S1* and *S2* differ, ALL_OR_NOTHING_GET returns the contents of *S3* as the value of the all-or-nothing sector. Figure 9.6 shows this almost correct pseudocode.

Let's explore how this implementation behaves on a system crash. Suppose that at some previous time a record has been correctly stored in an all-or-nothing sector (in other words, all three copies are identical), and someone now updates it by calling ALL_OR_NOTHING_PUT. The goal is that even if a failure occurs in the middle of the update, a later reader can always be ensured of getting some complete, consistent version of the record by invoking ALL_OR_NOTHING_GET.

Suppose that ALMOST_ALL_OR_NOTHING_PUT were interrupted by a system crash some time before it finishes writing sector *S2*, and thus corrupts either *S1* or *S2*. In that case,

```
1   procedure ALMOST_ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)
2       CAREFUL_PUT (data, all_or_nothing_sector.S1)
3       CAREFUL_PUT (data, all_or_nothing_sector.S2)          // Commit point.
4       CAREFUL_PUT (data, all_or_nothing_sector.S3)

5   procedure ALL_OR_NOTHING_GET (reference data, all_or_nothing_sector)
6       CAREFUL_GET (data1, all_or_nothing_sector.S1)
7       CAREFUL_GET (data2, all_or_nothing_sector.S2)
8       CAREFUL_GET (data3, all_or_nothing_sector.S3)
9       if data1 = data2 then data ← data1                    // Return new value.
10      else data ← data3                                     // Return old value.
```

图9.6

ALMOST_ALL_OR_NOTHING_PUT 和 ALL_OR_NOTHING_GET 的算法。

系统检测到故障后，会重置所有处理器寄存器和易失性内存，清除所有活动线程并重新启动。最多只有一个磁盘扇区会受损。

我们的问题现在简化为提供“全有或全无”属性：目标是创建 *all-or-nothing disk storage*，它保证要么完全且正确地更改扇区上的数据，要么让未来的读者看起来似乎根本没有触碰过它。这里有一个简单但效率稍低的方案，利用虚拟化技术：为每个需要具备“全有或全无”属性的数据扇区分配三个物理磁盘扇区，分别标识为*S1*、*S2*和*S3.* 。这三个物理扇区共同构成一个虚拟的“全有或全无扇区”。在系统中原先使用该磁盘扇区的每个位置，用这个由三元组{*S1*, *S2*, *S3*}标识的“全有或全无扇区”替换它。我们从一个近乎正确的“全有或全无”实现ALMOST_ALL_OR_NOTHING_PUT开始，发现其中的一个错误，然后修复这个错误，最终创建出正确的ALL_OR_NOTHING_PUT。

当要求写入数据时，ALMOST_ALL_OR_NOTHING_PUT 会按顺序在*S1*、*S2*和*S3,*上各写入三次，每次等待前一次写入完成，这样即使系统崩溃，三个扇区中也只有一个会受到影响。读取数据时，ALL_OR_NOTHING_GET会读取所有三个扇区并比较它们的内容。如果*S1*和*S2*的内容一致，ALL_OR_NOTHING_GET会将该值作为全有或全无扇区的值返回。若*S1*与*S2*存在差异，ALL_OR_NOTHING_GET则返回*S3*的内容作为全有或全无扇区的值。图9.6展示了这段近乎正确的伪代码。

让我们探讨这一实现在系统崩溃时的表现。假设在之前的某个时刻，一条记录已被正确存储在一个全有或全无扇区中（换言之，所有三个副本完全相同），现在有人通过调用ALL_OR_NOTHING_PUT来更新它。目标是即使更新过程中发生故障，后续的读取者通过调用ALL_OR_NOTHING_GET也能始终确保获得某个完整、一致的记录版本。

假设ALMOST_ALL_OR_NOTHING_PUT在完成写入扇区*S2*之前的某个时刻被系统崩溃中断，从而损坏了*S1*或*S2*。在这种情况下，

```
1  procedure ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)
2      CHECK_AND_REPAIR (all_or_nothing_sector)
3      ALMOST_ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)

4  procedure CHECK_AND_REPAIR (all_or_nothing_sector) // Ensure copies match.
5      CAREFUL_GET (data1, all_or_nothing_sector.S1)
6      CAREFUL_GET (data2, all_or_nothing_sector.S2)
7      CAREFUL_GET (data3, all_or_nothing_sector.S3)
8      if (data1 = data2) and (data2 = data3) return     // State 1 or 7, no repair
9      if (data1 = data2)
10         CAREFUL_PUT (data1, all_or_nothing_sector.S3) return    // State 5 or 6.
11     if (data2 = data3)
12         CAREFUL_PUT (data2, all_or_nothing_sector.S1) return    // State 2 or 3.
13     CAREFUL_PUT (data1, all_or_nothing_sector.S2)      // State 4, go to state 5
14     CAREFUL_PUT (data1, all_or_nothing_sector.S3)      // State 5, go to state 7
```

FIGURE 9.7

Algorithms for ALL_OR_NOTHING_PUT and CHECK_AND_REPAIR.

when ALL_OR_NOTHING_GET reads sectors *S1* and *S2*, they will have different values, and it is not clear which one to trust. Because the system is fail-fast, sector *S3* would not yet have been touched by ALMOST_ALL_OR_NOTHING_PUT, so it still contains the previous value. Returning the value found in *S3* thus has the desired effect of ALMOST_ALL_OR_NOTHING_PUT having done nothing.

Now, suppose that ALMOST_ALL_OR_NOTHING_PUT were interrupted by a system crash some time after successfully writing sector *S2*. In that case, the crash may have corrupted *S3*, but *S1* and *S2* both contain the newly updated value. ALL_OR_NOTHING_GET returns the value of *S1*, thus providing the desired effect of ALMOST_ALL_OR_NOTHING_PUT having completed its job.

So what's wrong with this design? ALMOST_ALL_OR_NOTHING_PUT assumes that all three copies are identical when it starts. But a previous failure can violate that assumption. Suppose that ALMOST_ALL_OR_NOTHING_PUT is interrupted while writing *S3*. The next thread to call ALL_OR_NOTHING_GET finds *data1* = *data2*, so it uses *data1*, as expected. The new thread then calls ALMOST_ALL_OR_NOTHING_PUT, but is interrupted while writing *S2*. Now, *S1* doesn't equal *S2*, so the next call to ALMOST_ALL_OR_NOTHING_PUT returns the damaged *S3*.

The fix for this bug is for ALL_OR_NOTHING_PUT to guarantee that the three sectors be identical before updating. It can provide this guarantee by invoking a procedure named CHECK_AND_REPAIR as in Figure 9.7. CHECK_AND_REPAIR simply compares the three copies and, if they are not identical, it forces them to be identical. To see how this works, assume that someone calls ALL_OR_NOTHING_PUT at a time when all three of the copies do contain identical values, which we designate as "old". Because ALL_OR_NOTHING_PUT writes "new"

```
1   procedure ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)
2       CHECK_AND_REPAIR (all_or_nothing_sector)
3       ALMOST_ALL_OR_NOTHING_PUT (data, all_or_nothing_sector)

4   procedure CHECK_AND_REPAIR (all_or_nothing_sector) // Ensure copies match.
5       CAREFUL_GET (data1, all_or_nothing_sector.S1)
6       CAREFUL_GET (data2, all_or_nothing_sector.S2)
7       CAREFUL_GET (data3, all_or_nothing_sector.S3)
8       if (data1 = data2) and (data2 = data3) return     // State 1 or 7, no repair
9       if (data1 = data2)
10          CAREFUL_PUT (data1, all_or_nothing_sector.S3) return    // State 5 or 6.
11      if (data2 = data3)
12          CAREFUL_PUT (data2, all_or_nothing_sector.S1) return    // State 2 or 3.
13      CAREFUL_PUT (data1, all_or_nothing_sector.S2)       // State 4, go to state 5
14      CAREFUL_PUT (data1, all_or_nothing_sector.S3)       // State 5, go to state 7
```

图9.7

ALL_OR_NOTHING_PUT 和 CHECK_AND_REPAIR 的算法。

当ALL_OR_NOTHING_GET读取扇区*S1*和*S2*时，它们的值会有所不同，此时不清楚该信任哪一个。由于系统采用快速失败机制，扇区*S3*尚未被ALMOST_ALL_OR_NOTHING_PUT触及，因此它仍保留着之前的值。返回*S3*中找到的值，就能达到ALMOST_ALL_OR_NOTHING_PUT 未执行任何操作的预期效果。

现在，假设ALMOST_ALL_OR_NOTHING_PUT 在成功写入扇区*S2*后不久因系统崩溃而中断。在这种情况下，崩溃可能已损坏*S3*，但*S1*和*S2*均包含新更新的值。ALL_OR_NOTHING_GET 返回*S1*的值，从而实现了ALMOST_ALL_OR_NOTHING_PUT已完成其工作的预期效果。

那么这个设计有什么问题呢？ ALMOST_ALL_OR_NOTHING_PUT假设开始时所有三个副本都是相同的。但之前的故障可能会违反这一假设。假设ALMOST_ALL_OR_NOTHING_PUT在写入*S3*时被中断。下一个调用ALL_OR_NOTHING_GET的线程会发现*data1=data2*，因此它会按预期使用*data1*。接着，新线程调用ALMOST_ALL_OR_NOTHING_PUT，但在写入*S2*时被中断。现在，*S1*不等于*S2*，所以下一次调用ALMOST_ALL_OR_NOTHING_PUT会返回损坏的*S3*。

针对此错误的修复方案是让ALL_OR_NOTHING_PUT 确保在更新前三个扇区完全相同。它可以通过调用名为CHECK_AND_REPAIR的过程来提供这一保证，如图9.7所示。CHECK_AND_REPAIR 只需比较三个副本，如果它们不一致，就会强制使其一致。T要理解其工作原理，假设有人在三个副本当前都包含相同值（我们称之为“旧”值）时调用ALL_OR_NOTHING_PUT 。由于ALL_OR_NOTHING_PUT会写入“新”值

values into *S1*, *S2*, and *S3* one at a time and in order, even if there is a crash, at the next call to ALL_OR_NOTHING_PUT there are only seven possible data states for CHECK_AND_REPAIR to consider:

data state:	1	2	3	4	5	6	7
sector *S1*	old	bad	new	new	new	new	new
sector *S2*	old	old	old	bad	new	new	new
sector *S3*	old	old	old	old	old	bad	new

The way to read this table is as follows: if all three sectors *S1*, *S2*, and *S3* contain the "old" value, the data is in state 1. Now, if CHECK_AND_REPAIR discovers that all three copies are identical (line 8 in Figure 9.7), the data is in state 1 or state 7 so CHECK_AND_REPAIR simply returns. Failing that test, if the copies in sectors *S1* and *S2* are identical (line 9), the data must be in state 5 or state 6, so CHECK_AND_REPAIR forces sector *S3* to match and returns (line 10). If the copies in sectors *S2* and *S3* are identical the data must be in state 2 or state 3 (line 11), so CHECK_AND_REPAIR forces sector *S1* to match and returns (line 12). The only remaining possibility is that the data is in state 4, in which case sector *S2* is surely bad, but sector *S1* contains a new value and sector *S3* contains an old one. The choice of which to use is arbitrary; as shown the procedure copies the new value in sector *S1* to both sectors *S2* and *S3*.

What if a failure occurs while running CHECK_AND_REPAIR? That procedure systematically drives the state either forward from state 4 toward state 7, or backward from state 3 toward state 1. If CHECK_AND_REPAIR is itself interrupted by another system crash, rerunning it will continue from the point at which the previous attempt left off.

We can make several observations about the algorithm implemented by ALL_OR_NOTHING_GET and ALL_OR_NOTHING_PUT:

1. This all-or-nothing atomicity algorithm assumes that only one thread at a time tries to execute either ALL_OR_NOTHING_GET or ALL_OR_NOTHING_PUT. This algorithm implements all-or-nothing atomicity but not before-or-after atomicity.

2. CHECK_AND_REPAIR is *idempotent*. That means that a thread can start the procedure, execute any number of its steps, be interrupted by a crash, and go back to the beginning again any number of times with the same ultimate result, as far as a later call to ALL_OR_NOTHING_GET is concerned.

3. The completion of the CAREFUL_PUT on line *3* of ALMOST_ALL_OR_NOTHING_PUT, marked "commit point," exposes the new data to future ALL_OR_NOTHING_GET actions. Until that step begins execution, a call to ALL_OR_NOTHING_GET sees the old data. After line *3* completes, a call to ALL_OR_NOTHING_GET sees the new data.

4. Although the algorithm writes three replicas of the data, the primary reason for the replicas is not to provide durability as described in Section 8.5. Instead, the reason for writing three replicas, one at a time and in a particular order, is to ensure observance at all times and under all failure scenarios of the *golden rule of atomicity*, which is the subject of the next section.

将值依次且按顺序填入*S1*、*S2*和*S3*，即使发生崩溃，在下次调用ALL_OR_NOTHING_PUT时，CHECK_AND_REPAIR只需考虑七种可能的数据状态：

data state:	1	2	3	4	5	6	7
sector *S1*	old	bad	new	new	new	new	new
sector *S2*	old	old	old	bad	new	new	new
sector *S3*	old	old	old	old	old	bad	new

阅读此表的方法如下：如果*S1*、*S2*和*S3*三个扇区均包含“旧”值，则数据处于状态1。此时，若CHECK_AND_REPAIR发现三个副本完全相同（图9.7第8行），则数据处于状态1或状态7，因此CHECK_AND_REPAIR直接返回。若该测试未通过，而*S1*和*S2*扇区的副本相同（第9行），则数据必然处于状态5或状态6，于是CHECK_AND_REPAIR强制*S3*扇区与之匹配并返回（第10行）。若*S2*和*S3*扇区的副本相同，则数据必处于状态2或状态3（第11行），因此CHECK_AND_REPAIR强制*S1*扇区与之匹配并返回（第12行）。唯一剩下的可能是数据处于状态4，此时*S2*扇区必定损坏，但*S1*扇区包含新值而*S3*扇区包含旧值。选择使用哪个值是任意的；如所示流程，将*S1*扇区的新值复制到*S2*和*S3*两个扇区。

如果在运行CHECK_AND_REPAIR时发生故障怎么办？该程序系统性地推动状态从状态4向前推进到状态7，或从状态3向后回退到状态1。如果CHECK_AND_REPAIR本身被另一次系统崩溃中断，重新运行它将从上一次尝试中断的地方继续。

我们可以对ALL_OR_NOTHING_GET和ALL_OR_NOTHING_PUT实现的算法做出几点观察：

1. 这一全有或全无原子性算法假设同一时间仅有一个线程尝试执行ALL_OR_NOTHING_GET或ALL_OR_NOTHING_PUT。该算法实现了全有或全无原子性，但未达成先后原子性。

2. CHECK_AND_REPAIR是*idempotent*。这意味着一个线程可以启动该过程，执行任意数量的步骤，被崩溃中断，然后无论多少次返回到起点，只要后续调用ALL_OR_NOTHING_GET，最终结果都是相同的。

3. 在ALMOST_ALL_OR_NOTHING_PUT的第*3*行完成标记为“提交点”的CAREFUL_PUT操作后，新数据将对未来的ALL_OR_NOTHING_GET操作可见。在该步骤开始执行之前，调用ALL_OR_NOTHING_GET将看到旧数据。而在第*3*行完成后，调用ALL_OR_NOTHING_GET则会看到新数据。

4. 尽管该算法会写入数据的三个副本，但副本的主要目的并非如第9.5节所述提供持久性保障。实际上，之所以依次按特定顺序写入三个副本，是为了在任何时刻及所有故障场景下都能确保遵守*golden rule of atomicity*原则，这一原则将是下一节讨论的主题。

There are several ways of implementing all-or-nothing disk sectors. Near the end of Chapter 8[on-line] we introduced a fault tolerance model for decay events that did not mask system crashes, and applied the technique known as RAID to mask decay to produce durable storage. Here we started with a slightly different fault tolerance model that omits decay, and we devised techniques to mask system crashes and produce all-or-nothing storage. What we really should do is start with a fault tolerance model that considers both system crashes and decay, and devise storage that is both all-or-nothing and durable. Such a model, devised by Xerox Corporation researchers Butler Lampson and Howard Sturgis, is the subject of Section 9.7, together with the more elaborate recovery algorithms it requires. That model has the additional feature that it needs only two physical sectors for each all-or-nothing sector.

9.2.2 Systematic Atomicity: Commit and the Golden Rule

The example of ALL_OR_NOTHING_PUT and ALL_OR_NOTHING_GET demonstrates an interesting special case of all-or-nothing atomicity, but it offers little guidance on how to systematically create a more general all-or-nothing action. From the example, our calendar program now has a tool that allows writing individual sectors with the all-or-nothing property, but that is not the same as safely adding an event to a calendar, since adding an event probably requires rearranging a data structure, which in turn may involve writing more than one disk sector. We could do a series of ALL_OR_NOTHING_PUTs to the several sectors, to ensure that each sector is itself written in an all-or-nothing fashion, but a crash that occurs after writing one and before writing the next would leave the overall calendar addition in a partly-done state. To make the entire calendar addition action all-or-nothing we need a generalization.

Ideally, one might like to be able to take any arbitrary sequence of instructions in a program, surround that sequence with some sort of **begin** and **end** statements as in Figure 9.8, and expect that the language compilers and operating system will perform some magic that makes the surrounded sequence into an all-or-nothing action. Unfortunately, no one knows how to do that. But we can come close, if the programmer is willing to make a modest concession to the requirements of all-or-nothing atomicity. This concession is expressed in the form of a discipline on the constituent steps of the all-or-nothing action.

The discipline starts by identifying some single step of the sequence as the *commit point*. The all-or-nothing action is thus divided into two phases, a *pre-commit phase* and a *post-commit phase*, as suggested by Figure 9.9. During the pre-commit phase, the disciplining rule of design is that no matter what happens, it must be possible to back out of this all-or-nothing action in a way that leaves no trace. During the post-commit phase the disciplining rule of design is that no matter what happens, the action must run to the end successfully. Thus an all-or-nothing action can have only two outcomes. If the all-or-nothing action starts and then, without reaching the commit point, backs out, we say that it *aborts*. If the all-or-nothing action passes the commit point, we say that it *commits*.

实现全有或全无磁盘扇区有几种方法。在第8章[在线]接近尾声处，我们介绍了一种针对衰变事件的容错模型，该模型不掩盖系统崩溃，并应用了称为RAID的技术来屏蔽衰变，以实现持久存储。这里，我们从一个略为不同的容错模型出发，该模型忽略了衰变，转而设计了掩盖系统崩溃并实现全有或全无存储的技术。实际上，我们应当从同时考虑系统崩溃和衰变的容错模型入手，设计出既全有或全无又持久的存储系统。由Xerox公司的研究人员Butler Lampson和Howard Sturgis提出的这样一个模型，连同其所需的更复杂恢复算法，构成了第9.7节的主题。该模型还有一个额外特点，即每个全有或全无扇区仅需两个物理扇区。

9.2.2 系统性原子性：提交与黄金法则

ALL_OR_NOTHING_PUT 和ALL_OR_NOTHING_GET 的例子展示了全有或全无原子性（all-or-nothing atomicity）的一个有趣特例，但它并未提供如何系统化构建更通用全有或全无操作的指导。通过该示例，我们的日历程序现在拥有了一种工具，能够以全有或全无属性写入单个扇区，但这与安全地向日历添加事件并不相同——因为添加事件可能需要重组数据结构，进而涉及写入多个磁盘扇区。我们可以对多个扇区执行一系列ALL_OR_NOTHING_PUT操作，确保每个扇区自身以全有或全无方式写入，但若在写入一个扇区后、下一个扇区前发生崩溃，仍会导致整个日历添加操作处于部分完成状态。要实现整个日历添加操作的全有或全无特性，我们需要一种通用化的解决方案。

理想情况下，人们可能希望能够在程序中选取任意指令序列，像图9.8所示那样用某种**begin**和**end** 语句将其包围，并期待语言编译器和操作系统施展某种魔法，使被包围的序列成为全有或全无的操作。遗憾的是，目前无人知晓如何实现这一点。但如果程序员愿意为满足全有或全无的原子性要求做出适度让步，我们就能接近这一目标。这种让步体现在对全有或全无操作的组成步骤施加一套规范约束。

该学科首先将序列中的某个单一步骤识别为*commit point*。因此，如图9.9所示，全有或全无操作被划分为两个阶段：*pre-commit phase*和*post-commit phase*。在预提交阶段，设计的约束规则是无论发生什么情况，都必须能够以不留痕迹的方式回退这一全有或全无操作。而在后提交阶段，设计的约束规则则是无论发生什么情况，操作都必须成功执行到底。因此，全有或全无操作只有两种可能结果。若操作启动后未达提交点即回退，则称其*aborts*；若操作通过了提交点，则称其*commits*。

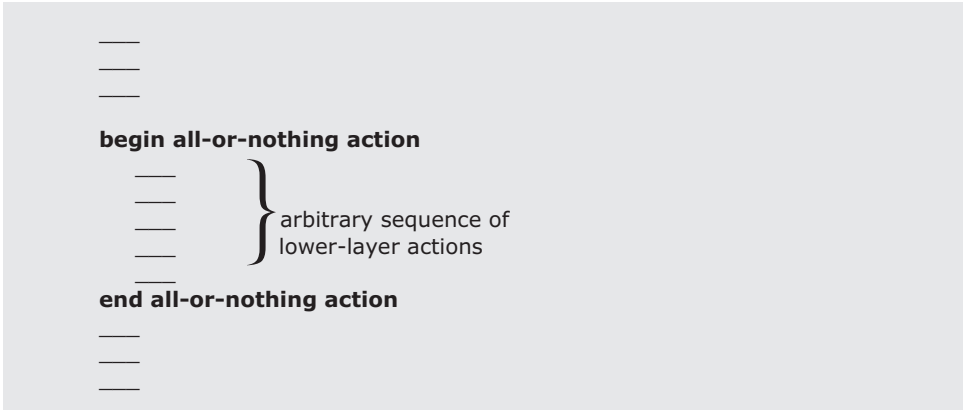

FIGURE 9.8

Imaginary semantics for painless programming of all-or-nothing actions.

We can make several observations about the restrictions of the pre-commit phase. The pre-commit phase must identify all the resources needed to complete the all-or-nothing action, and establish their availability. The names of data should be bound, permissions should be checked, the pages to be read or written should be in memory, removable media should be mounted, stack space must be allocated, etc. In other words, all the steps needed to anticipate the severe run-to-the-end-without-faltering requirement of the post-commit phase should be completed during the pre-commit phase. In addition, the pre-commit phase must maintain the ability to abort at any instant. Any changes that the pre-commit phase makes to the state of the system must be undoable in case this all-or-nothing action aborts. Usually, this requirement means that shared

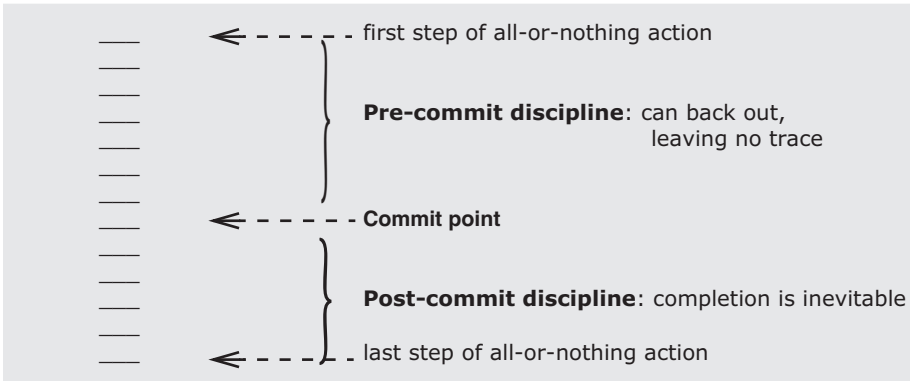

FIGURE 9.9

The commit point of an all-or-nothing action.

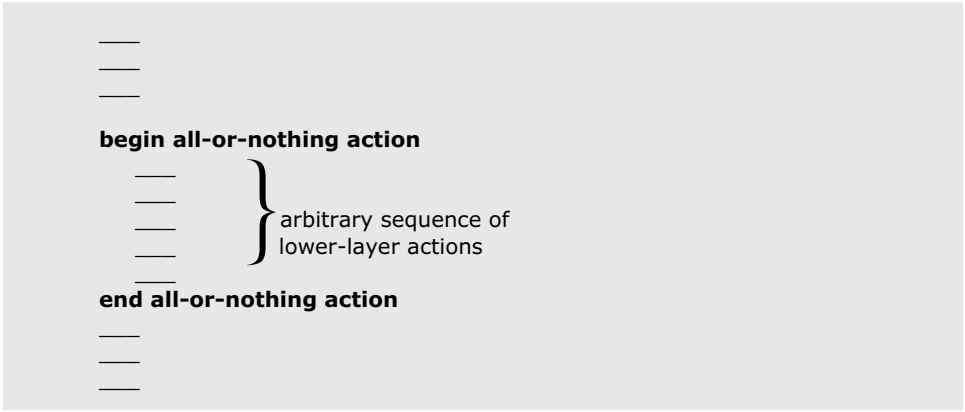

图9.8

我全有或全无行为无痛编程的虚拟语义　　离子。

我们可以对预提交阶段的限制做出几点观察。预提交阶段必须识别完成全有或全无操作所需的所有资源，并确认它们的可用性。数据名称应被绑定，权限需得到检查，待读取或写入的页面应在内存中，可移动介质需挂载，栈空间必须分配等。换言之，预提交阶段应完成所有必要步骤，以预见并满足后提交阶段“一鼓作气执行到底”的严苛要求。此外，预提交阶段必须保持随时中止的能力。若此全有或全无操作中止，预提交阶段对系统状态的任何更改都必须是可撤销的。通常，这一要求意味着共享的

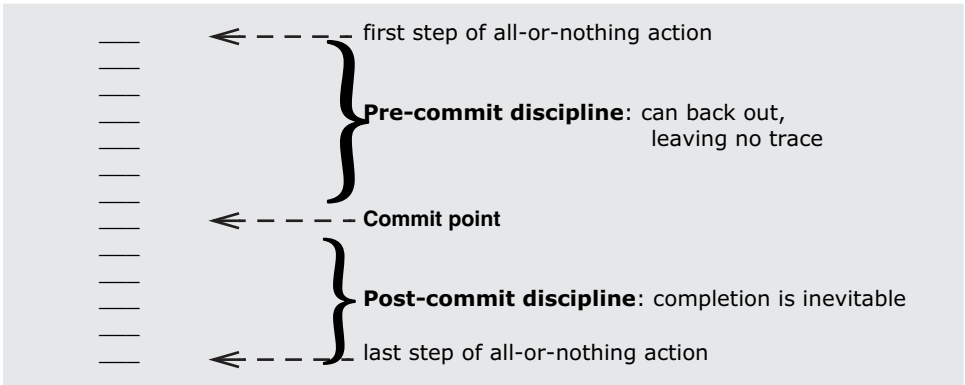

图9.9

全有或全无行动的提交点。

resources, once reserved, cannot be released until the commit point is passed. The reason is that if an all-or-nothing action releases a shared resource, some other, concurrent thread may capture that resource. If the resource is needed in order to undo some effect of the all-or-nothing action, releasing the resource is tantamount to abandoning the ability to abort. Finally, the reversibility requirement means that the all-or-nothing action should not do anything externally visible, for example printing a check or firing a missile, prior to the commit point. (It is possible, though more complicated, to be slightly less restrictive. Sidebar 9.3 explores that possibility.)

In contrast, the post-commit phase can expose results, it can release reserved resources that are no longer needed, and it can perform externally visible actions such as printing a check, opening a cash drawer, or drilling a hole. But it cannot try to acquire additional resources because an attempt to acquire might fail, and the post-commit phase is not permitted the luxury of failure. The post-commit phase must confine itself to finishing just the activities that were planned during the pre-commit phase.

It might appear that if a system fails before the post-commit phase completes, all hope is lost, so the only way to ensure all-or-nothing atomicity is to always make the commit step the last step of the all-or-nothing action. Often, that is the simplest way to ensure all-or-nothing atomicity, but the requirement is not actually that stringent. An important feature of the post-commit phase is that it is hidden inside the layer that implements the all-or-nothing action, so a scheme that ensures that the post-commit phase completes *after* a system failure is acceptable, so long as this delay is hidden from the invoking layer. Some all-or-nothing atomicity schemes thus involve a guarantee that a cleanup procedure will be invoked following every system failure, or as a prelude to the next use of the data, before anyone in a higher layer gets a chance to discover that anything went wrong. This idea should sound familiar: the implementation of ALL_OR_NOTHING_PUT in Figure 9.7 used this approach, by always running the cleanup procedure named CHECK_AND_REPAIR before updating the data.

A popular technique for achieving all-or-nothing atomicity is called the *shadow copy*. It is used by text editors, compilers, calendar management programs, and other programs that modify existing files, to ensure that following a system failure the user does not end up with data that is damaged or that contains only some of the intended changes:

- Pre-commit: Create a complete duplicate working copy of the file that is to be modified. Then, make all changes to the working copy.

Sidebar 9.3: Cascaded aborts *(Temporary) sweeping simplification.* In this initial discussioin of commit points, we are intentionally avoiding a more complex and harder-to-design possibility. Some systems allow other, concurrent activities to see pending results, and they may even allow externally visible actions before commit. Those systems must therefore be prepared to track down and abort those concurrent activities (this tracking down is called *cascaded abort*) or perform *compensating* external actions (e.g., send a letter requesting return of the check or apologizing for the missile firing). The discussion of layers and multiple sites in Chapter 10[online] introduces a simple version of cascaded abort.

资源一旦被预留，在通过提交点之前无法释放。原因在于，如果一个全有或全无操作释放了共享资源，其他并发线程可能会占用该资源。若撤销全有或全无操作的某些效果需要该资源，释放资源就等于放弃了中止操作的能力。最后，可逆性要求意味着全有或全无操作在提交点之前不应执行任何外部可见的操作，例如打印支票或发射导弹。（虽然限制可以稍宽松些，但这会使情况更复杂。侧边栏9.3探讨了这种可能性。）

相比之下，提交后阶段可以公开结果，释放不再需要的预留资源，并执行外部可见的操作，如打印支票、打开现金抽屉或钻孔。但它不能尝试获取额外资源，因为获取尝试可能会失败，而提交后阶段不允许有失败的余地。提交后阶段必须仅限于完成那些在预提交阶段就已规划好的活动。

乍看之下，若系统在提交后阶段完成前发生故障，似乎所有努力都将付诸东流，因此确保全有或全无原子性的唯一方法就是始终将提交步骤作为整个动作的最后一步。虽然这通常是保证原子性最简单的方式，但实际要求并非如此严苛。提交后阶段的关键特性在于它被封装在实现全有或全无动作的层级内部，因此只要这种延迟对调用层透明，允许提交后阶段在系统故障后*after*完成的方案也是可行的。某些全有或全无原子性方案会通过以下机制实现：确保每次系统故障后都会调用清理程序，或在下次使用数据前预先执行清理——这一切都发生在更高层级有机会察觉异常之前。这个思路应当似曾相识：图9.7中ALL_OR_NOTHING_PUT的实现就采用了该策略，它始终通过名为CHECK_AND_REPAIR的清理程序完成数据更新前的预处理工作。

一种实现全有或全无原子性的流行技术被称为*shadow copy*。它被文本编辑器、编译器、日历管理程序以及其他修改现有文件的程序所采用，以确保在系统故障后，用户不会得到损坏的数据或仅包含部分预期更改的数据：

- 预提交：创建待修改文件的完整工作副本。然后，对该工作副本进行所有更改。

边栏9.3：级联中止*(Temporary) sweeping simplification*。在关于提交点的初步讨论中，我们有意避开了一种更复杂且设计难度更大的可能性。某些系统允许其他并发活动查看待定的结果，甚至可能在提交前允许执行外部可见的操作。因此，这些系统必须准备好追踪并中止这些并发活动（这种追踪称为*cascaded abort*），或执行*compensating*外部操作（例如，发送信件要求退回支票或为导弹发射道歉）。第10章[在线]中关于层次结构和多站点的讨论引入了级联中止的一个简单版本。

- Commit point: Carefully exchange the working copy with the original. Typically this step is bootstrapped, using a lower-layer RENAME entry point of the file system that provides certain atomic-like guarantees such as the ones described for the UNIX version of RENAME in Section 2.5.8.
- Post-commit: Release the space that was occupied by the original.

The ALL_OR_NOTHING_PUT algorithm of Figure 9.7 can be seen as a particular example of the shadow copy strategy, which itself is a particular example of the general pre-commit/post-commit discipline. The commit point occurs at the instant when the new value of *S2* is successfully written to the disk. During the pre-commit phase, while ALL_OR_NOTHING_PUT is checking over the three sectors and writing the shadow copy *S1*, a crash will leave no trace of that activity (that is, no trace that can be discovered by a later caller of ALL_OR_NOTHING_GET). The post-commit phase of ALL_OR_NOTHING_PUT consists of writing *S3*.

From these examples we can extract an important design principle:

The golden rule of atomicity

Never modify the only copy!

In order for a composite action to be all-or-nothing, there must be some way of reversing the effect of each of its pre-commit phase component actions, so that if the action does not commit it is possible to back out. As we continue to explore implementations of all-or-nothing atomicity, we will notice that correct implementations always reduce at the end to making a shadow copy. The reason is that structure ensures that the implementation follows the golden rule.

9.2.3 Systematic All-or-Nothing Atomicity: Version Histories

This section develops a scheme to provide all-or-nothing atomicity in the general case of a program that modifies arbitrary data structures. It will be easy to see why the scheme is correct, but the mechanics can interfere with performance. Section 9.3 of this chapter then introduces a variation on the scheme that requires more thought to see why it is correct, but that allows higher-performance implementations. As before, we concentrate for the moment on all-or-nothing atomicity. While some aspects of before-or-after atomicity will also emerge, we leave a systematic treatment of that topic for discussion in Sections 9.4 and 9.5 of this chapter. Thus the model to keep in mind in this section is that only a single thread is running. If the system crashes, after a restart the original thread is gone—recall from Chapter 8[on-line] the *sweeping simplification* that threads are included in the volatile state that is lost on a crash and only durable state survives. After the crash, a new, different thread comes along and attempts to look at the data. The goal is that the new thread should always find that the all-or-nothing action that was in progress at the time of the crash either never started or completed successfully.

•提交点：谨慎地将工作副本与原始版本进行交换。通常，这一步是通过文件系统的低层RENAME入口点引导完成的，该入口点提供类似原子性的保证，例如第2.5.8节中针对UNIX版RENAME 所描述的那些保证。

•提交后：释放原占用的空间。

图9.7中的ALL_OR_NOTHING_PUT算法可视为影子拷贝策略的一个具体实例，而该策略本身又是通用预提交/后提交规范的一个特例。提交点发生在*S2*新值成功写入磁盘的瞬间。在预提交阶段，当ALL_OR_NOTHING_PUT检查三个扇区并写入影子拷贝*S1*时，系统崩溃不会留下该活动的任何痕迹（即后续调用ALL_OR_NOTHING_GET时无法发现的痕迹）。ALL_OR_NOTHING_PUT的后提交阶段则包含写入*S3*的操作。

从这些例子中我们可以提炼出一个重要的设计原则：

原子性的黄金法则

Never modify the only copy!

为了使一个复合动作具有“全有或全无”的特性，必须存在某种方式能够逆转其预提交阶段各组成动作的效果，这样若动作未能提交，就有办法回退。随着我们继续探索“全有或全无”原子性的实现方式，会发现正确的实现最终总是归结为创建影子副本。原因在于这种结构确保了实现遵循黄金法则。

9.2.3 系统性全有或全无原子性：版本历史

本节提出了一种方案，旨在为修改任意数据结构的通用程序提供全有或全无的原子性保障。该方案的正确性显而易见，但其机制可能影响性能。本章第9.3节将介绍该方案的变体，其正确性需要更多思考来理解，但允许更高性能的实现。如前所述，我们目前专注于全有或全无的原子性。虽然前后原子性的某些方面也会显现，但我们把对该主题的系统性讨论留到本章第9.4和9.5节。因此，本节需牢记的模型是仅有一个线程在运行。若系统崩溃，重启后原线程将消失——回忆第8章[在线]所述*sweeping simplification*，线程属于易失状态，崩溃时丢失，仅持久状态得以保留。崩溃后，一个新的不同线程会尝试查看数据。目标是确保新线程总能发现崩溃时正在进行全有或全无操作要么从未开始，要么已成功完成。

In looking at the general case, a fundamental difficulty emerges: random-access memory and disk usually appear to the programmer as a set of named, shared, and rewritable storage cells, called *cell storage*. Cell storage has semantics that are actually quite hard to make all-or-nothing because the act of storing destroys old data, thus potentially violating the golden rule of atomicity. If the all-or-nothing action later aborts, the old value is irretrievably gone; at best it can only be reconstructed from information kept elsewhere. In addition, storing data reveals it to the view of later threads, whether or not the all-or-nothing action that stored the value reached its commit point. If the all-or-nothing action happens to have exactly one output value, then writing that value into cell storage can be the mechanism of committing, and there is no problem. But if the result is supposed to consist of several output values, all of which should be exposed simultaneously, it is harder to see how to construct the all-or-nothing action. Once the first output value is stored, the computation of the remaining outputs has to be successful; there is no going back. If the system fails and we have not been careful, a later thread may see some old and some new values.

These limitations of cell storage did not plague the shopkeepers of Padua, who in the 14th century invented double-entry bookkeeping. Their storage medium was leaves of paper in bound books and they made new entries with quill pens. They never erased or even crossed out entries that were in error; when they made a mistake they made another entry that reversed the mistake, thus leaving a complete history of their actions, errors, and corrections in the book. It wasn't until the 1950's, when programmers began to automate bookkeeping systems, that the notion of overwriting data emerged. Up until that time, if a bookkeeper collapsed and died while making an entry, it was always possible for someone else to seamlessly take over the books. This observation about the robustness of paper systems suggests that there is a form of the golden rule of atomicity that might allow one to be systematic: never erase anything.

Examining the shadow copy technique used by the text editor provides a second useful idea. The essence of the mechanism that allows a text editor to make several changes to a file, yet not reveal any of the changes until it is ready, is this: the only way another prospective reader of a file can reach it is by name. Until commit time the editor works on a copy of the file that is either not yet named or has a unique name not known outside the thread, so the modified copy is effectively invisible. Renaming the new version is the step that makes the entire set of updates simultaneously visible to later readers.

These two observations suggest that all-or-nothing actions would be better served by a model of storage that behaves differently from cell storage: instead of a model in which a store operation overwrites old data, we instead create a new, tentative version of the data, such that the tentative version remains invisible to any reader outside this all-or-nothing action until the action commits. We can provide such semantics, even though we start with traditional cell memory, by interposing a layer between the cell storage and the program that reads and writes data. This layer implements what is known as *journal storage*. The basic idea of journal storage is straightforward: we associate with every named variable not a single cell, but a list of cells in non-volatile storage; the values in the list represent the history of the variable. Figure 9.10 illustrates. Whenever any action

在探讨一般情况时，一个根本性难题浮现出来：随机存取存储器和磁盘对程序员而言，通常表现为一组被命名、共享且可重写的存储单元，称为*cell storage*。存储单元具有这样的语义特性——实际上极难实现全有或全无的操作，因为存储行为会破坏旧数据，从而可能违反原子性的黄金法则。若后续全有或全无操作中止，旧值将不可挽回地丢失；至多只能通过其他位置保留的信息进行重建。此外，数据一旦存储就会暴露给后续线程查看，无论存储该值的全有或全无操作是否已达到提交点。若全有或全无操作恰好只产生一个输出值，那么将该值写入存储单元可作为提交机制，此时不存在问题。但如果结果应由多个输出值构成，且这些值需同时对外可见，则较难构建全有或全无操作。一旦首个输出值被存储，剩余输出的计算就必须成功；此时已无退路。若系统发生故障而我们未加防范，后续线程可能会看到部分旧值与部分新值混杂的情况。

这些细胞存储的局限并未困扰帕多瓦的商人们，他们在14世纪发明了复式记账法。他们的存储介质是装订成册的纸张，用羽毛笔记录新条目。他们从不擦除甚至划掉错误的记录；一旦出错，就另作一笔反向冲销的条目，从而在账簿中完整保留操作、错误及更正的痕迹。直到20世纪50年代，当程序员开始将记账系统自动化时，覆写数据的概念才出现。在此之前，若记账员在记录时猝死，其他人总能无缝接管账簿。这种纸质系统稳健性的观察启示我们，或许存在一种原子性黄金法则的形式，让人能够系统化操作：永不抹去任何记录。

研究文本编辑器所使用的影子副本技术，为我们提供了第二个实用的思路。该机制的核心在于：文本编辑器能够对文件进行多次修改，却能在完成前不显露任何改动，其奥秘在于——其他潜在文件读取者只能通过文件名访问文件。在提交之前，编辑器操作的文件副本要么尚未命名，要么拥有线程外不可知的唯一名称，因此修改后的副本实际上处于不可见状态。通过重命名新版本这一步骤，所有更新内容便能同时呈现给后续的读取者。

这两点观察表明，全有或全无操作更适合采用一种与单元存储行为相异的存储模型：不同于旧数据被存储操作直接覆盖的模型，我们转而创建数据的新暂定版本，使得该版本在此全有或全无操作提交前，对任何外部读取者均不可见。即便从传统单元内存出发，我们仍能通过介入一个位于单元存储与读写数据的程序之间的中间层，来提供此类语义。该层实现了所谓的*journal storage*。日志存储的基本理念直截了当：我们为每个命名变量关联的并非单一单元，而是非易失性存储中的单元列表；列表中的值记录了变量的历史。图9.10对此进行了说明。每当任一操作

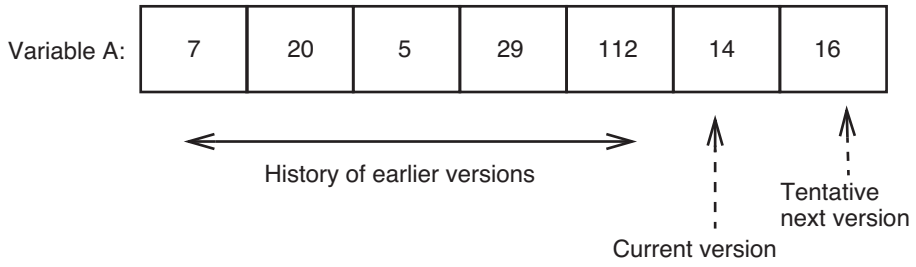

FIGURE 9.10

Version history of a variable in journal storage.

proposes to write a new value into the variable, the journal storage manager appends the prospective new value to the end of the list. Clearly this approach, being history-preserving, offers some hope of being helpful because if an all-or-nothing action aborts, one can imagine a systematic way to locate and discard all of the new versions it wrote. Moreover, we can tell the journal storage manager to expect to receive tentative values, but to ignore them unless the all-or-nothing action that created them commits. The basic mechanism to accomplish such an expectation is quite simple; the journal storage manager should make a note, next to each new version, of the identity of the all-or-nothing action that created it. Then, at any later time, it can discover the status of the tentative version by inquiring whether or not the all-or-nothing action ever committed.

Figure 9.11 illustrates the overall structure of such a journal storage system, implemented as a layer that hides a cell storage system. (To reduce clutter, this journal storage system omits calls to create new and delete old variables.) In this particular model, we assign to the journal storage manager most of the job of providing tools for programming all-or-nothing actions. Thus the implementer of a prospective all-or-nothing action should begin that action by invoking the journal storage manager entry NEW_ACTION, and later complete the action by invoking either COMMIT or ABORT. If, in addition, actions perform all reads and writes of data by invoking the journal storage manager's READ_CURRENT_VALUE and WRITE_NEW_VALUE entries, our hope is that the result will automatically be all-or-nothing with no further concern of the implementer.

How could this automatic all-or-nothing atomicity work? The first step is that the journal storage manager, when called at NEW_ACTION, should assign a nonce identifier to the prospective all-or-nothing action, and create, in non-volatile cell storage, a record of this new identifier and the state of the new all-or-nothing action. This record is called an *outcome record*; it begins its existence in the state PENDING; depending on the outcome it should eventually move to one of the states COMMITTED or ABORTED, as suggested by Figure 9.12. No other state transitions are possible, except to discard the outcome record once

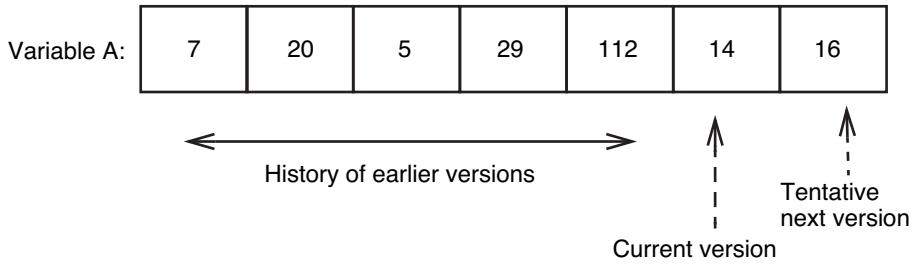

图9.10

变量在日志存储中的版本历史。

提议向变量写入新值时，日志存储管理器会将预期的新值追加到列表末尾。显然，这种保留历史记录的方法有望发挥作用，因为如果一个全有或全无操作中止，可以设想一种系统化的方法来定位并丢弃它所写入的所有新版本。此外，我们可以指示日志存储管理器准备接收暂定值，但只有在创建它们的全有或全无操作提交后，才予以采纳。实现这种预期的基础机制相当简单：日志存储管理器应在每个新版本旁边记录创建它的全有或全无操作的标识。之后，它随时可以通过查询该全有或全无操作是否已提交，来发现暂定版本的状态。

图9.11展示了此类日志存储系统的整体结构，该系统作为隐藏单元存储系统的抽象层实现。（为简化图示，该日志存储系统省略了创建新变量与删除旧变量的调用接口。）在此特定模型中，我们将提供全有或全无编程工具的主要职责赋予日志存储管理器。因此，全有或全无动作的实现者应通过调用日志存储管理器入口NEW_ACTION来启动该动作，随后通过调用COMMIT或ABORT来完成动作。若动作执行过程中所有数据读写操作均通过调用日志存储管理器的READ_CURRENT_VALUE和WRITE_NEW_VALUE入口完成，我们期望该动作将自动具备全有或全无特性，无需实现者额外关注。

这种自动的全有或全无原子性如何实现？第一步是，日志存储管理器在NEW_ACTION被调用时，应为预期的全有或全无操作分配一个唯一标识符，并在非易失性存储单元中创建一条记录，记载这一新标识符及新全有或全无操作的状态。该记录被称为*outcome record*；它初始存在于PENDING状态；根据操作结果，它最终应迁移至COMMITTED或ABORTED状态之一，如图9.12所示。除最终废弃结果记录外，不允许其他状态转换。

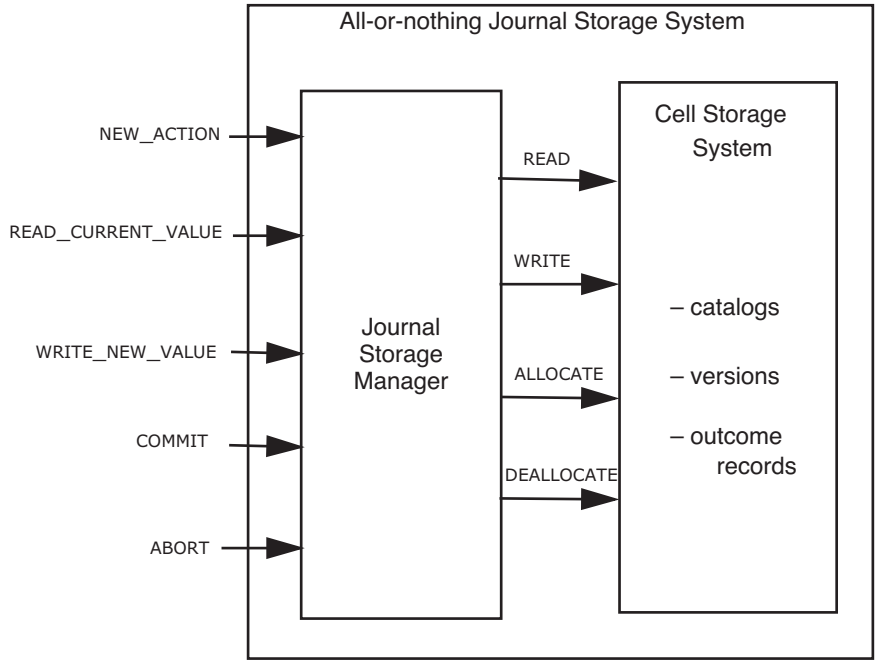

FIGURE 9.11

Interface to and internal organization of an all-or-nothing storage system based on version histories and journal storage.

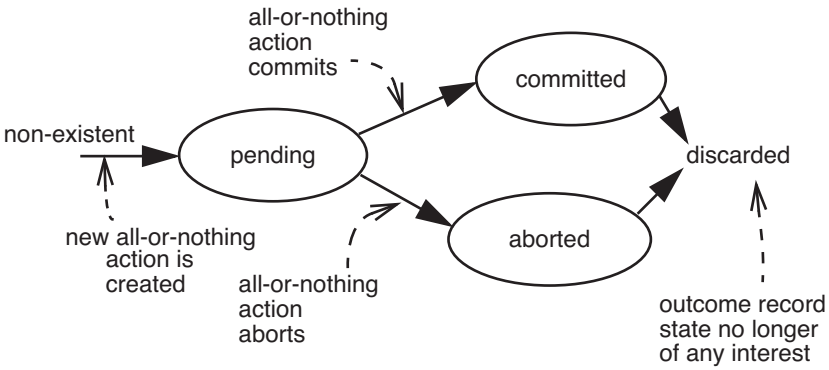

FIGURE 9.12

The allowed state transitions of an outcome record.

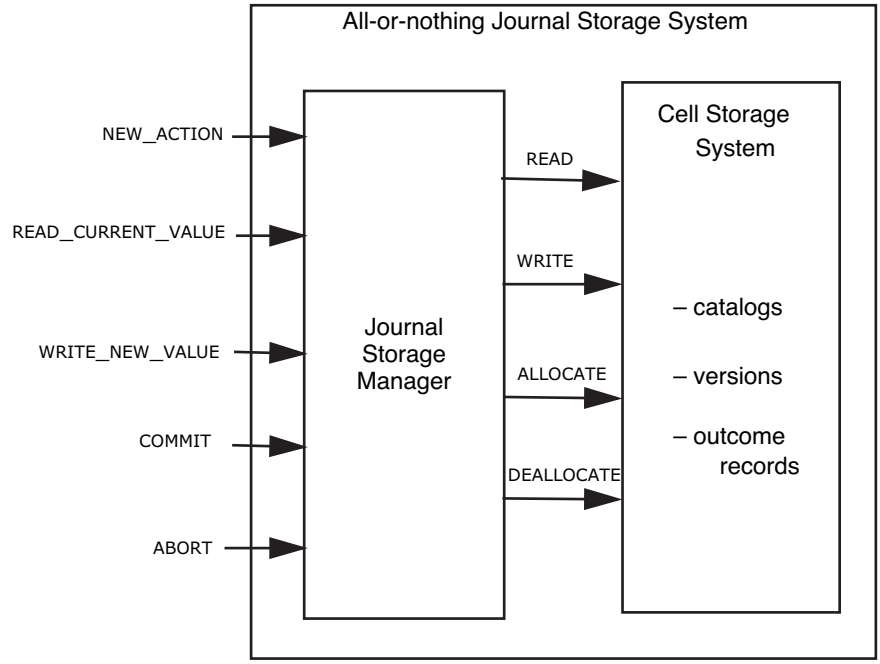

图9.11

基于版本历史和日志存储的全有或全无存储系统的接口与内部组织结构。

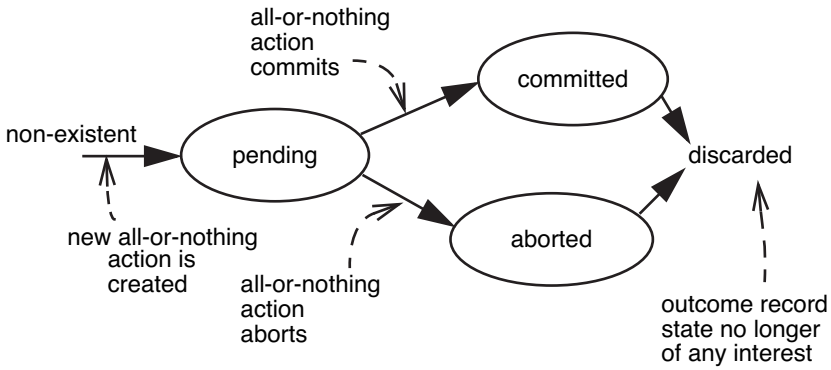

图9.12

The allowed state transitions of an outcome record.

```
1  procedure NEW_ACTION ()
2      id ← NEW_OUTCOME_RECORD ()
3      id.outcome_record.state ← PENDING
4      return id

5  procedure COMMIT (reference id)
6      id.outcome_record.state ← COMMITTED

7  procedure ABORT (reference id)
8      id.outcome_record.state ← ABORTED
```

FIGURE 9.13

The procedures NEW_ACTION, COMMIT, and ABORT.

there is no further interest in its state. Figure 9.13 illustrates implementations of the three procedures NEW_ACTION, COMMIT, and ABORT.

When an all-or-nothing action calls the journal storage manager to write a new version of some data object, that action supplies the identifier of the data object, a tentative new value for the new version, and the identifier of the all-or-nothing action. The journal storage manager calls on the lower-level storage management system to allocate in nonvolatile cell storage enough space to contain the new version; it places in the newly allocated cell storage the new data value and the identifier of the all-or-nothing action. Thus the journal storage manager creates a version history as illustrated in Figure 9.14. Now,

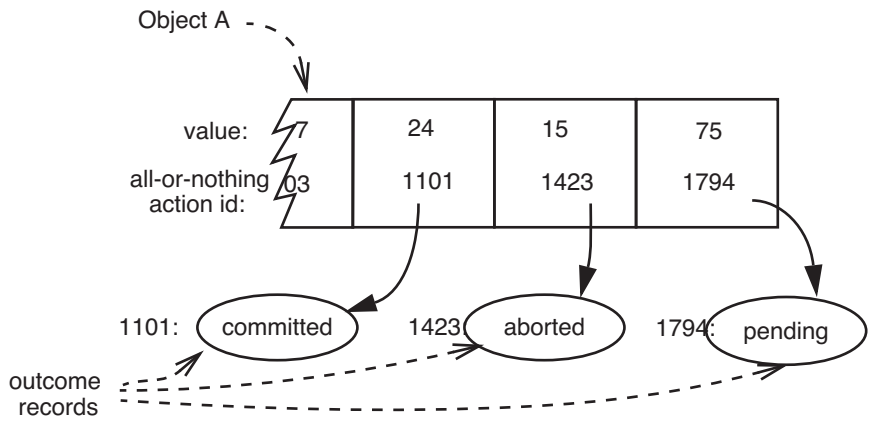

FIGURE 9.14

Portion of a version history, with outcome records. Some thread has recently called WRITE_NEW_VALUE specifying *data_id = A*, *new_value = 75*, and *client_id = 1794*. A caller to READ_CURRENT_VALUE will read the value 24 for A.

```
1  procedure NEW_ACTION ()
2      id ← NEW_OUTCOME_RECORD ()
3      id.outcome_record.state ← PENDING
4      return id

5  procedure COMMIT (reference id)
6      id.outcome_record.state ← COMMITTED

7  procedure ABORT (reference id)
8      id.outcome_record.state ← ABORTED
```

图9.13

程序 NEW_ACTION、 COMMIT 和 ABORT。

对其状态已无进一步兴趣。图9.13展示了三个程序NEW_ACTION、 COMMIT和ABORT的实现。

当一个全有或全无操作调用日志存储管理器写入某个数据对象的新版本时，该操作需提供数据对象的标识符、新版本的暂定值以及全有或全无操作的标识符。日志存储管理器会调用底层存储管理系统，在非易失性单元存储中分配足够空间以容纳新版本；随后将新数据值及全有或全无操作的标识符存入新分配的存储单元。如此，日志存储管理器便创建了如图9.14所示的版本历史。此时，

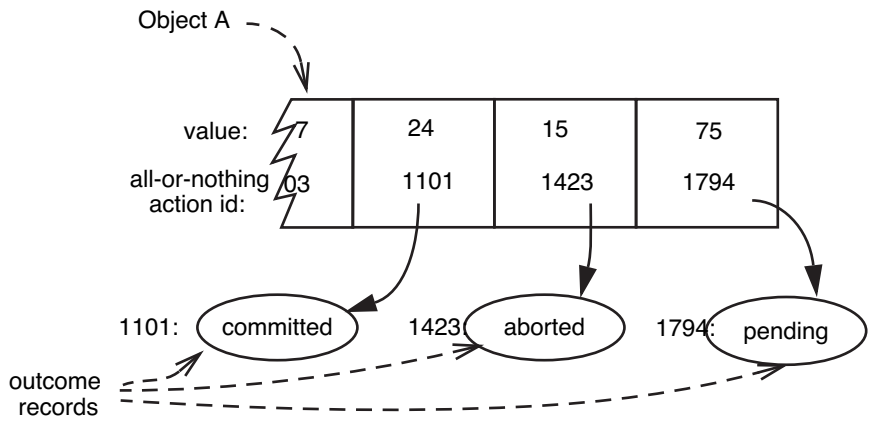

图9.14

版本历史的一部分，附有结果记录。某个线程最近调用了WRITE_NEW_VALUE，指定了 *data_id = A*、*new_value = 75*和*client_id = 1794*。调用READ_CURRENT_VALUE的调用者将读取到A的值为24。

```
1  procedure READ_CURRENT_VALUE (data_id, caller_id)
2      starting at end of data_id repeat until beginning
3          v ← previous version of data_id        // Get next older version
4          a ← v.action_id  // Identify the action a that created it
5          s ← a.outcome_record.state             // Check action a's outcome record
6          if s = COMMITTED then
7              return v.value
8          else skip v                            // Continue backward search
9      signal ("Tried to read an uninitialized variable!")

10 procedure WRITE_NEW_VALUE (reference data_id, new_value, caller_id)
11     if caller_id.outcome_record.state = PENDING
12         append new version v to data_id
13         v.value ← new_value
14         v.action_id ← caller_id
                   else signal ("Tried to write outside of an all-or-nothing action!")
```

FIGURE 9.15

Algorithms followed by READ_CURRENT_VALUE and WRITE_NEW_VALUE. The parameter *caller_id* is the action identifier returned by NEW_ACTION. In this version, only WRITE_NEW_VALUE uses *caller_id*. Later, READ_CURRENT_VALUE will also use it.

when someone proposes to read a data value by calling READ_CURRENT_VALUE, the journal storage manager can review the version history, starting with the latest version and return the value in the most recent committed version. By inspecting the outcome records, the journal storage manager can ignore those versions that were written by all-or-nothing actions that aborted or that never committed.

The procedures READ_CURRENT_VALUE and WRITE_NEW_VALUE thus follow the algorithms of Figure 9.15. The important property of this pair of algorithms is that if the current all-or-nothing action is somehow derailed before it reaches its call to COMMIT, the new version it has created is invisible to invokers of READ_CURRENT_VALUE. (They are also invisible to the all-or-nothing action that wrote them. Since it is sometimes convenient for an all-or-nothing action to read something that it has tentatively written, a different procedure, named READ_MY_PENDING_VALUE, identical to READ_CURRENT_VALUE except for a different test on line 6, could do that.) Moreover if, for example, all-or-nothing action 99 crashes while partway through changing the values of nineteen different data objects, all nineteen changes would be invisible to later invokers of READ_CURRENT_VALUE. If all-or-nothing action 99 does reach its call to COMMIT, that call commits the entire set of changes simultaneously and atomically, at the instant that it changes the outcome record from PENDING to COMMITTED. Pending versions would also be invisible to any concurrent action that reads data with READ_CURRENT_VALUE, a feature that will prove useful when we introduce concurrent threads and discuss before-or-after atomicity, but for the moment our only

```
1  procedure READ_CURRENT_VALUE (data_id, caller_id)
2      starting at end of data_id repeat until beginning
3          v ← previous version of data_id        // Get next older version
4          a ← v.action_id  // Identify the action a that created it
5          s ← a.outcome_record.state             // Check action a's outcome record
6          if s = COMMITTED then
7              return v.value
8          else skip v                            // Continue backward search
9      signal ("Tried to read an uninitialized variable!")

10 procedure WRITE_NEW_VALUE (reference data_id, new_value, caller_id)
11     if caller_id.outcome_record.state = PENDING
12         append new version v to data_id
13         v.value ← new_value
14         v.action_id ← caller_id
                   else signal ("Tried to write outside of an all-or-nothing action!")
```

图9.15

遵循READ_CURRENT_VALUE和WRITE_NEW_VALUE的算法。参数*caller_id*是由NEW_ACTION返回的动作标识符。在此版本中，仅WRITE_NEW_VALUE使用了*caller_id*。后续，READ_CURRENT_VALUE也将采用它。

当有人提议通过调用READ_CURRENT_VALUE,来读取数据值时，日志存储管理器可以审查版本历史，从最新版本开始，并返回最近提交版本中的值。通过检查结果记录，日志存储管理器可以忽略那些由全有或全无动作写入的版本，这些动作要么已中止，要么从未提交。

过程READ_CURRENT_VALUE和WRITE_NEW_VALUE因此遵循图9.15中的算法。这对算法的重要特性在于：若当前全有或全无操作在抵达COMMIT调用前被意外中断，其所创建的新版本对READ_CURRENT_VALUE调用者不可见。（这些版本对写入它们的全有或全无操作本身也不可见。由于全有或全无操作有时需要读取其暂存写入的内容，可通过另一个名为READ_MY_PENDING_VALUE的过程实现——该过程除第6行,的判定条件不同外与READ_CURRENT_VALUE 完全相同。）此外，例如当全有或全无操作99在修改十九个数据对象的过程中崩溃时，所有十九项修改对后续READ_CURRENT_VALUE调用者均不可见。若全有或全无操作99成功执行至COMMIT调用，该调用将在将结果记录从PENDING 改为COMMITTED的瞬间，以原子方式同时提交全部变更集。待定版本对任何使用READ_CURRENT_VALUE读取数据的并发操作同样不可见，这一特性将在引入并发线程讨论前后原子性时发挥重要作用，但目前我们仅需...

```
1  procedure TRANSFER (reference debit_account, reference credit_account,
                                                                amount)
2      my_id ← NEW_ACTION ()
3      xvalue ← READ_CURRENT_VALUE (debit_account, my_id)
4      xvalue ← xvalue - amount
5      WRITE_NEW_VALUE (debit_account, xvalue, my_id)
6      yvalue ← READ_CURRENT_VALUE (credit_account, my_id)
7      yvalue ← yvalue + amount
8      WRITE_NEW_VALUE (credit_account, yvalue, my_id)
9      if xvalue > 0 then
10         COMMIT (my_id)
11     else
12         ABORT (my_id)
13         signal("Negative transfers are not allowed.")
```

FIGURE 9.16

An all-or-nothing TRANSFER procedure, based on journal storage. (This program assumes that it is the only running thread. Making the transfer procedure a before-or-after action because other threads might be updating the same accounts concurrently requires additional mechanism that is discussed later in this chapter.)

concern is that a system crash may prevent the current thread from committing or aborting, and we want to make sure that a later thread doesn't encounter partial results. As in the case of the calendar manager of Section 9.2.1, we assume that when a crash occurs, any all-or-nothing action that was in progress at the time was being supervised by some outside agent who realizes that a crash has occurred, uses READ_CURRENT_VALUE to find out what happened and if necessary initiates a replacement all-or-nothing action.

Figure 9.16 shows the TRANSFER procedure of Section 9.1.5 reprogrammed as an all-or-nothing (but not, for the moment, before-or-after) action using the version history mechanism. This implementation of TRANSFER is more elaborate than the earlier one—it tests to see whether or not the account to be debited has enough funds to cover the transfer and if not it aborts the action. The order of steps in the transfer procedure is remarkably unconstrained by any consideration other than calculating the correct answer. The reading of *credit_account*, for example, could casually be moved to any point between NEW_ACTION and the place where *yvalue* is recalculated. We conclude that the journal storage system has made the pre-commit discipline much less onerous than we might have expected.

There is still one loose end: it is essential that updates to a version history and changes to an outcome record be all-or-nothing. That is, if the system fails while the thread is inside WRITE_NEW_VALUE, adjusting structures to append a new version, or inside COMMIT while updating the outcome record, the cell being written must not be muddled; it must either stay as it was before the crash or change to the intended new value. The solution is to design all modifications to the internal structures of journal storage so that they can

```
1   procedure TRANSFER (reference debit_account, reference credit_account,
                                                                  amount)
2       my_id ← NEW_ACTION ()
3       xvalue ← READ_CURRENT_VALUE (debit_account, my_id)
4       xvalue ← xvalue - amount
5       WRITE_NEW_VALUE (debit_account, xvalue, my_id)
6       yvalue ← READ_CURRENT_VALUE (credit_account, my_id)
7       yvalue ← yvalue + amount
8       WRITE_NEW_VALUE (credit_account, yvalue, my_id)
9       if xvalue > 0 then
10          COMMIT (my_id)
11      else
12          ABORT (my_id)
13          signal("Negative transfers are not allowed.")
```

图9.16

一种基于日志存储的全有或全无TRANSFER过程。（此程序假设它是唯一运行的线程。将转账过程设为原子操作，因为其他线程可能同时更新同一账户，这需要本章后续讨论的额外机制。）

担忧在于系统崩溃可能会阻止当前线程提交或中止操作，而我们希望确保后续线程不会遇到部分结果。如同第9.2.1节中的日历管理器案例，我们假设当崩溃发生时，任何正在进行中的全有或全无操作都由某个外部代理监管，该代理意识到崩溃已发生，会利用{v*}来查明情况，并在必要时启动替代的全有或全无操作。

图9.16展示了第9.1.5节中的TRANSFER流程通过版本历史机制被重新编程为一个全有或全无（但暂时不考虑先后顺序）的操作。这一TRANSFER的实现比之前的版本更为精细——它会检查待扣款账户是否有足够资金完成转账，若不足则中止操作。转账流程中的步骤顺序除计算正确结果外，几乎不受任何其他因素约束。例如，读取*credit_account*的操作可以随意调整到NEW_ACTION与重新计算*yvalue*之间的任意位置。我们由此得出结论：日志存储系统使得预提交规则的执行负担比预期轻得多。

还有一个未解决的问题：对版本历史的更新和结果记录的变更必须是全有或全无的。也就是说，如果系统在线程处于WRITE_NEW_VALUE阶段（调整结构以追加新版本）或COMMIT 阶段（更新结果记录）时发生故障，正在写入的单元格状态必须保持清晰；它要么保持崩溃前的原状，要么完全变更为预期的新值。解决方案在于设计日志存储内部结构的所有修改操作，确保它们能够

be done by overwriting a single cell. For example, suppose that the name of a variable that has a version history refers to a cell that contains the address of the newest version, and that versions are linked from the newest version backwards, by address references. Adding a version consists of allocating space for a new version, reading the current address of the prior version, writing that address in the backward link field of the new version, and then updating the descriptor with the address of the new version. That last update can be done by overwriting a single cell. Similarly, updating an outcome record to change it from PENDING to COMMITTED can be done by overwriting a single cell.

As a first bootstrapping step, we have reduced the general problem of creating all-or-nothing actions to the specific problem of doing an all-or-nothing overwrite of one cell. As the remaining bootstrapping step, recall that we already know two ways to do a single-cell all-or-nothing overwrite: apply the ALL_OR_NOTHING_PUT procedure of Figure 9.7. (If there is concurrency, updates to the internal structures of the version history also need before-or-after atomicity. Section 9.4 will explore methods of providing it.)

9.2.4 How Version Histories are Used

The careful reader will note two possibly puzzling things about the version history scheme just described. Both will become less puzzling when we discuss concurrency and before-or-after atomicity in Section 9.4 of this chapter:

1. Because READ_CURRENT_VALUE skips over any version belonging to another all-or-nothing action whose OUTCOME record is not COMMITTED, it isn't really necessary to change the OUTCOME record when an all-or-nothing action aborts; the record could just remain in the PENDING state indefinitely. However, when we introduce concurrency, we will find that a pending action may prevent other threads from reading variables for which the pending action created a new version, so it will become important to distinguish aborted actions from those that really are still pending.

2. As we have defined READ_CURRENT_VALUE, versions older than the most recent committed version are inaccessible and they might just as well be discarded. Discarding could be accomplished either as an additional step in the journal storage manager, or as part of a separate garbage collection activity. Alternatively, those older versions may be useful as an historical record, known as an *archive*, with the addition of timestamps on commit records and procedures that can locate and return old values created at specified times in the past. For this reason, a version history system is sometimes called a *temporal database* or is said to provide *time domain addressing*. The banking industry abounds in requirements that make use of history information, such as reporting a consistent sum of balances in all bank accounts, paying interest on the fifteenth on balances as of the first of the month, or calculating the average balance last month. Another reason for not discarding old versions immediately will emerge when we discuss concurrency and

可以通过覆盖单个单元格来完成。例如，假设一个拥有版本历史的变量名指向一个包含最新版本地址的单元格，且各版本通过地址引用从最新版本向后链接。添加新版本包括为新版本分配空间、读取前一个版本的当前地址、将该地址写入新版本的后向链接字段，然后更新描述符为新版本的地址。最后这一更新操作只需覆盖单个单元格即可实现。同理，将结果记录从PENDING更新为COMMITTED 也只需覆盖单个单元格。

作为第一个引导步骤，我们已经将创建全有或全无动作的普遍问题简化为对单个单元格. 进行全有或全无覆盖的具体问题。在剩余的引导步骤中，回顾已知两种实现单单元格全有或全无覆盖的方法：应用图9.7中的ALL_OR_NOTHING_PUT流程（若存在并发操作，版本历史内部结构的更新同样需满足前或后原子性。第9.4节将探讨实现这一保障的方法。）

9.2.4 版本历史的使用方式

细心的读者可能会注意到，关于刚刚描述的版本历史方案，有两处可能令人困惑的地方。当我们讨论本章第9.4节中的并发性和前后原子性时，这两点将变得不那么令人费解：

1. 由于READ_CURRENT_VALUE 会跳过属于任何其他全有或全无动作的版本（这些动作的OUTCOME记录未处于COMMITTED状态），因此当全有或全无动作中止时，实际上并不需要更改OUTCOME记录；该记录可以无限期地保持在PENDING状态。然而，当我们引入并发性时，会发现一个待处理的动作可能会阻止其他线程读取该动作创建了新版本的变量，因此区分已中止的动作与真正仍处于待处理状态的动作将变得至关重要。

2. 正如我们定义了READ_CURRENT_VALUE，比最近提交版本更早的版本将无法访问，它们完全可以被丢弃。丢弃操作可以作为日志存储管理器的一个额外步骤来实现，或者作为独立垃圾回收活动的一部分。另一种情况是，这些旧版本可能作为历史记录（称为*archive*）而具有价值，只需在提交记录上添加时间戳，并配备能够定位并返回过去特定时间创建的旧值的程序即可。因此，版本历史系统有时被称为*temporal database*，或被认为提供了*time domain addressing*。银行业中存在大量利用历史信息的需求，例如报告所有银行账户余额的一致总和、在每月15日根据当月1日的余额支付利息，或计算上月的平均余额。当我们讨论并发性时，还会出现另一个不立即丢弃旧版本的原因。

before-or-after atomicity: concurrent threads may, for correctness, need to read old versions even after new versions have been created and committed.

Direct implementation of a version history raises concerns about performance: rather than simply reading a named storage cell, one must instead make at least one indirect reference through a descriptor that locates the storage cell containing the current version. If the cell storage device is on a magnetic disk, this extra reference is a potential bottleneck, though it can be alleviated with a cache. A bottleneck that is harder to alleviate occurs on updates. Whenever an application writes a new value, the journal storage layer must allocate space in unused cell storage, write the new version, and update the version history descriptor so that future readers can find the new version. Several disk writes are likely to be required. These extra disk writes may be hidden inside the journal storage layer and with added cleverness may be delayed until commit and batched, but they still have a cost. When storage access delays are the performance bottleneck, extra accesses slow things down.

In consequence, version histories are used primarily in low-performance applications. One common example is found in revision management systems used to coordinate teams doing program development. A programmer "checks out" a group of files, makes changes, and then "checks in" the result. The check-out and check-in operations are all-or-nothing and check-in makes each changed file the latest version in a complete history of that file, in case a problem is discovered later. (The check-in operation also verifies that no one else changed the files while they were checked out, which catches some, but not all, coordination errors.) A second example is that some interactive applications such as word processors or image editing systems provide a "deep undo" feature, which allows a user who decides that his or her recent editing is misguided to step backwards to reach an earlier, satisfactory state. A third example appears in file systems that automatically create a new version every time any application opens an existing file for writing; when the application closes the file, the file system tags a number suffix to the name of the previous version of the file and moves the original name to the new version. These interfaces employ version histories because users find them easy to understand and they provide all-or-nothing atomicity in the face of both system failures and user mistakes. Most such applications also provide an archive that is useful for reference and that allows going back to a known good version.

Applications requiring high performance are a different story. They, too, require all-or-nothing atomicity, but they usually achieve it by applying a specialized technique called a *log*. Logs are our next topic.

9.3 All-or-Nothing Atomicity II: Pragmatics

Database management applications such as airline reservation systems or banking systems usually require high performance as well as all-or-nothing atomicity, so their designers use streamlined atomicity techniques. The foremost of these techniques sharply separates the reading and writing of data from the failure recovery mechanism.

前后原子性：为了正确性，并发线程可能需要在创建并提交新版本后仍读取旧版本。

直接实现版本历史会引发对性能的担忧：与简单地读取一个命名的存储单元不同，必须至少通过一个间接引用的描述符来定位包含当前版本的存储单元。如果存储设备是磁盘，这种额外的引用可能成为性能瓶颈，尽管可以通过缓存来缓解。更难以缓解的瓶颈出现在更新操作上。每当应用程序写入新值时，日志存储层必须在未使用的存储空间中分配位置，写入新版本，并更新版本历史描述符，以便后续读取者能找到新版本。这一过程可能需要多次磁盘写入。这些额外的磁盘写入可以被隐藏在日志存储层内部，通过巧妙设计延迟到提交时批量处理，但它们依然存在开销。当存储访问延迟成为性能瓶颈时，额外的访问操作会拖慢整体速度。

因此，版本历史主要应用于低性能场景。一个常见例子是用于协调程序开发团队的修订管理系统。程序员“签出”一组文件进行修改后“签入”结果。签出与签入操作具有原子性——签入操作会将每个被修改文件确立为该文件完整历史中的最新版本，以便后续发现问题时追溯（签入操作还会验证文件在签出期间未被他人修改，此举能捕获部分而非全部协作错误）。第二个例子体现在某些交互式应用中，如文字处理或图像编辑系统提供的“深度撤销”功能，允许用户发现近期编辑不当后回退至早期满意状态。第三个例子是某些文件系统会在应用程序以写入模式打开现有文件时自动创建新版本：当应用程序关闭文件时，系统会为旧版本文件名添加数字后缀，并将原名称赋予新版本。这些界面采用版本历史机制，因其易于用户理解，且能在系统故障和用户错误时提供原子性保障。多数此类应用还提供存档功能，既便于参考，也能回滚至已知稳定版本。

需要高性能的应用则是另一回事。它们同样要求全有或全无的原子性，但通常通过采用一种称为*log*的专门技术来实现。日志是我们接下来要讨论的主题。

9.3 全有或全无原子性II：实用考量

诸如航空订票系统或银行系统之类的数据库管理应用通常需要高性能以及全有或全无的原子性，因此其设计者采用简化的原子性技术。这些技术中最重要的一点是将数据的读写操作与故障恢复机制严格分离。

The idea is to minimize the number of storage accesses required for the most common activities (application reads and updates). The trade-off is that the number of storage accesses for rarely-performed activities (failure recovery, which one hopes is actually exercised only occasionally, if at all) may not be minimal. The technique is called *logging*. Logging is also used for purposes other than atomicity, several of which Sidebar 9.4 describes.

9.3.1 Atomicity Logs

The basic idea behind atomicity logging is to combine the all-or-nothing atomicity of journal storage with the speed of cell storage, by having the application twice record every change to data. The application first *logs* the change in journal storage, and then it *installs* the change in cell storage*. One might think that writing data twice must be more expensive than writing it just once into a version history, but the separation permits specialized optimizations that can make the overall system faster.

The first recording, to journal storage, is optimized for fast writing by creating a single, interleaved version history of all variables, known as a *log*. The information describing each data update forms a record that the application appends to the end of the log. Since there is only one log, a single pointer to the end of the log is all that is needed to find the place to append the record of a change of any variable in the system. If the log medium is magnetic disk, and the disk is used only for logging, and the disk storage management system allocates sectors contiguously, the disk seek arm will need to move only when a disk cylinder is full, thus eliminating most seek delays. As we will see, recovery does involve scanning the log, which is expensive, but recovery should be a rare event. Using a log is thus an example of following the hint to *optimize for the common case*.

The second recording, to cell storage, is optimized to make reading fast: the application installs by simply overwriting the previous cell storage record of that variable. The record kept in cell storage can be thought of as a cache that, for reading, bypasses the effort that would be otherwise be required to locate the latest version in the log. In addition, by not reading from the log the logging disk's seek arm can remain in position, ready for the next update. The two steps, LOG and INSTALL, become a different implementation of the WRITE_NEW_VALUE interface of Figure 9.11. Figure 9.17 illustrates this two-step implementation.

The underlying idea is that the log is the authoritative record of the outcome of the action. Cell storage is merely a reference copy; if it is lost, it can be reconstructed from the log. The purpose of installing a copy in cell storage is to make both logging and reading faster. By recording data twice, we obtain high performance in writing, high performance in reading, and all-or-nothing atomicity, all at the same time.

There are three common logging configurations, shown in Figure 9.18. In each of these three configurations, the log resides in non-volatile storage. For the *in-memory*

* A hardware architect would say "…it *graduates* the change to cell storage". This text, somewhat arbitrarily, chooses to use the database management term "install" .

其核心思想是最小化最常见操作（应用程序读取和更新）所需的存储访问次数。代价则是那些极少执行的操作（故障恢复，人们希望这类操作即便存在也应极少发生）可能无法实现最小化存储访问。该技术被称为*logging*。日志记录不仅用于保证原子性，还被用于其他目的，其中几个用途在侧边栏9.4中有所阐述。

9.3.1 原子性日志

原子性日志记录的基本思想是将日志存储的全有或全无原子性与单元存储的速度相结合，通过让应用程序对数据的每次变更进行两次记录来实现。应用程序首先在日志存储中*logs*变更，随后在单元存储中*installs* 变更*。有人可能会认为，将数据写入两次比仅写入版本历史一次成本更高，但这种分离允许专门的优化，从而可能使整个系统运行得更快。

首次记录至日志存储时，通过为所有变量创建单一交错版本历史（称为*log*，优化了快速写入性能。描述每次数据更新的信息形成一条记录，应用程序将其追加到日志末尾。由于仅存在单一日志，仅需一个指向日志末端的指针即可定位系统中任意变量变更记录的追加位置。若日志介质为磁盘，且该磁盘专用于日志记录，同时磁盘存储管理系统连续分配扇区，则磁盘寻道臂仅在柱面写满时才需移动，从而消除大多数寻道延迟。后续将看到，恢复过程确实涉及扫描日志（这一操作成本较高），但恢复应属罕见事件。因此，采用日志正是遵循*optimize for the common case*提示的范例。

第二次记录，即对单元存储的写入，经过优化以实现快速读取：应用程序只需覆盖该变量之前的单元存储记录即可完成安装。保存在单元存储中的记录可视为一种缓存，读取时无需费力在日志中定位最新版本。此外，由于不从日志读取，日志磁盘的寻道臂可保持原位，随时准备下一次更新。这两个步骤，LOG和INSTALL，成为图9.11中WRITE_NEW_VALUE接口的另一种实现方式。图9.17展示了这种两步实现方案。

其核心理念在于，日志是记录操作结果的权威依据。单元存储仅作为参考副本；若发生丢失，可从日志中重建。在单元存储中安装副本的目的，是为了同时提升日志记录和读取的速度。通过双重记录数据，我们得以在写入时获得高性能，在读取时同样高效，同时确保操作的原子性——要么全部完成，要么全部不执行。

有三种常见的日志配置，如图9.18所示。在这三种配置中，日志均驻留在非易失性存储介质中。对于*in-memory*

* A hardware architect would say "…it *graduates* the change to cell storage". This text, somewhat arbitrarily, chooses to use the database management term "install" .

Sidebar 9.4: The many uses of logs A log is an object whose primary usage method is to append a new record. Log implementations normally provide procedures to read entries from oldest to newest or in reverse order, but there is usually not any procedure for modifying previous entries. Logs are used for several quite distinct purposes, and this range of purposes sometimes gets confused in real-world designs and implementations. Here are some of the most common uses for logs:

1. *Atomicity log*. If one logs the component actions of an all-or-nothing action, together with sufficient before and after information, then a crash recovery procedure can undo (and thus roll back the effects of) all-or-nothing actions that didn't get a chance to complete, or finish all-or-nothing actions that committed but that didn't get a chance to record all of their effects.
2. *Archive log*. If the log is kept indefinitely, it becomes a place where old values of data and the sequence of actions taken by the system or its applications can be kept for review. There are many uses for archive information: watching for failure patterns, reviewing the actions of the system preceding and during a security breach, recovery from application-layer mistakes (e.g., a clerk incorrectly deleted an account), historical study, fraud control, and compliance with record-keeping requirements.
3. *Performance log*. Most mechanical storage media have much higher performance for sequential access than for random access. Since logs are written sequentially, they are ideally suited to such storage media. It is possible to take advantage of this match to the physical properties of the media by structuring data to be written in the form of a log. When combined with a cache that eliminates most disk reads, a performance log can provide a significant speed-up. As will be seen in the accompanying text, an atomicity log is usually also a performance log.
4. *Durability log*. If the log is stored on a non-volatile medium—say magnetic tape—that fails in ways and at times that are independent from the failures of the cell storage medium—which might be magnetic disk—then the copies of data in the log are replicas that can be used as backup in case of damage to the copies of the data in cell storage. This kind of log helps implement durable storage. Any log that uses a non-volatile medium, whether intended for atomicity, archiving or performance, typically also helps support durability.

It is essential to have these various purposes—all-or-nothing atomicity, archive, performance, and durable storage—distinct in one's mind when examining or designing a log implementation because they lead to different priorities among design trade-offs. When archive is the goal, low cost of the storage medium is usually more important than quick access because archive logs are large but, in practice, infrequently read. When durable storage is the goal, it may be important to use storage media with different physical properties, so that failure modes will be as independent as possible. When all-or-nothing atomicity or performance is the purpose, minimizing mechanical movement of the storage device becomes a high priority. Because of the competing objectives of different kinds of logs, as a general rule, it is usually a wise move to implement separate, dedicated logs for different functions.

侧边栏 9.4：日志的多种用途　日志是一种以追加新记录为主要使用方式的对象。日志实现通常提供按从旧到新或相反顺序读取条目的过程，但通常不提供修改先前条目的任何过程。日志被用于几种截然不同的目的，而这些用途的范围有时会在实际设计和实现中造成混淆。以下是日志最常见的一些用途：

1. *Atomicity log*。如果记录下一个全有或全无操作的组件动作，并附带足够的操作前后信息，那么崩溃恢复过程就能够撤销（从而回滚）那些未能完成的全有或全无操作的效果，或者完成那些已提交但未能记录其全部效果的全有或全无操作。

2. *Archive log*。如果日志被无限期保留，它便成为一个可以存储数据旧值以及系统或其应用程序所采取操作序列以供审查的地方。存档信息有多种用途：监测故障模式、审查系统在安全漏洞发生前后及期间的行为、从应用层错误中恢复（例如，职员误删账户）、历史研究、欺诈控制，以及满足记录保存的合规要求。

3. *Performance log*。大多数机械存储介质在顺序访问时的性能远高于随机访问。由于日志是按顺序写入的，它们非常适合这类存储介质。通过将待写入的数据结构化为日志形式，可以充分利用介质物理特性的这一匹配优势。当与一个能消除大多数磁盘读取操作的缓存结合使用时，性能日志可以带来显著的速度提升。正如附文所述，原子性日志通常同时也是性能日志。

4. *Durability log*。如果日志存储在非易失性介质上——例如磁带——其故障方式和时间与单元存储介质（可能是磁盘）的故障相互独立，那么日志中的数据副本可作为备份，在单元存储中的数据副本受损时使用。这类日志有助于实现持久化存储。任何使用非易失性介质的日志，无论其设计初衷是为了原子性、归档还是性能，通常也都有助于支持持久性。

在审视或设计日志实现时，务必清晰区分这些不同用途——全有或全无的原子性、归档、性能以及持久化存储——因为它们会导致设计权衡中的优先级差异。当目标是归档时，存储介质的低成本通常比快速访问更为重要，因为归档日志体积庞大但在实践中很少被读取。若以实现持久化存储为目标，则可能需要采用具有不同物理特性的存储介质，以确保故障模式尽可能独立。当追求全有或全无原子性或性能时，最大限度减少存储设备的机械运动就成为高度优先事项。鉴于各类日志存在相互冲突的目标，通常明智的做法是为不同功能实现独立专用的日志。

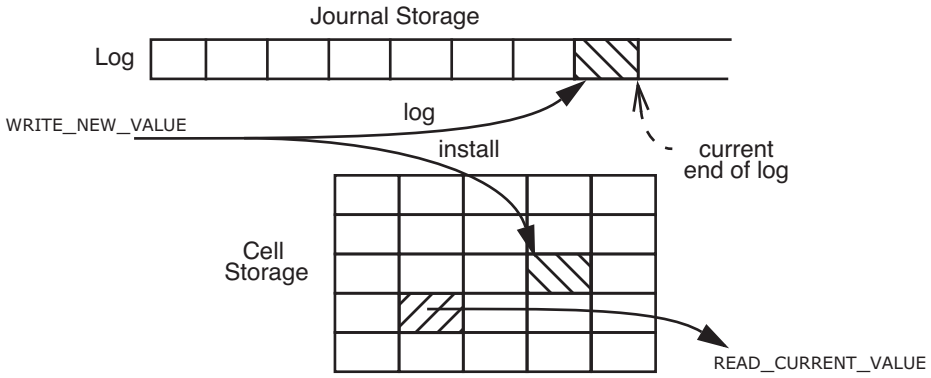

FIGURE 9.17

Logging for all-or-nothing atomicity. The application performs WRITE_NEW_VALUE by first appending a record of the new value to the log in journal storage, and then installing the new value in cell storage by overwriting. The application performs READ_CURRENT_VALUE by reading just from cell storage.

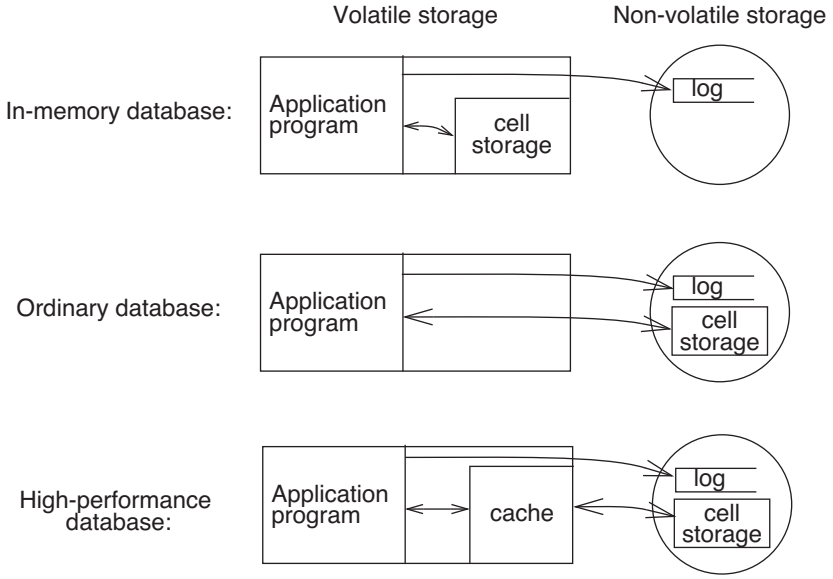

FIGURE 9.18

Three common logging configurations. Arrows show data flow as the application reads, logs, and installs data.

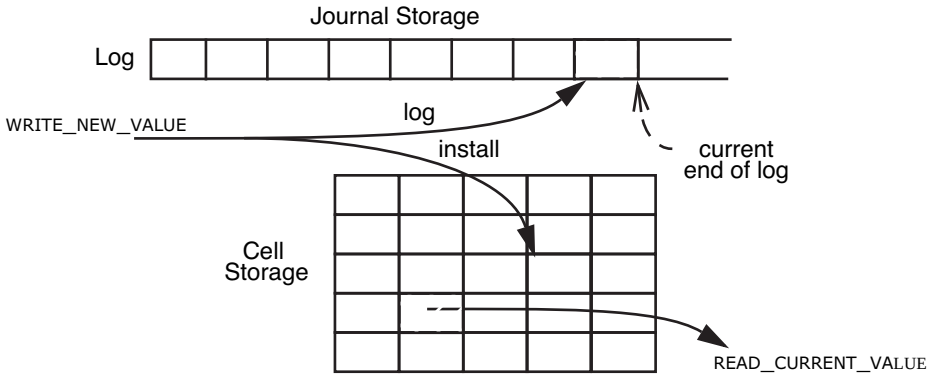

图9.17

记录以实现全有或全无的原子性。应用程序通过首先将新值的记录追加到日志存储中的日志，然后通过覆盖方式在单元存储中安装新值，来执行WRITE_NEW_VALUE。应用程序通过仅从单元存储读取，来执行READ_CURRENT_VALUE。

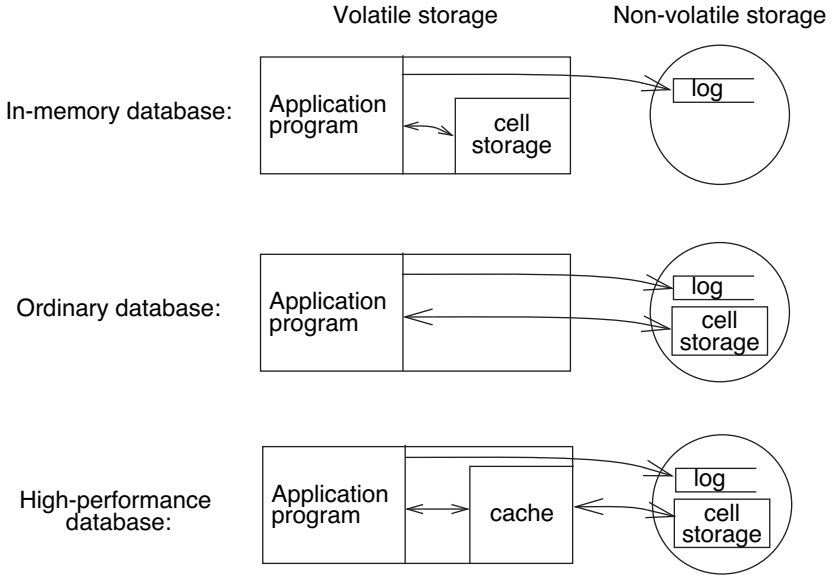

图9.18

三种常见的日志配置。箭头展示了应用程序读取、记录和安装数据时的数据流向。

database, cell storage resides entirely in some volatile storage medium. In the second common configuration, cell storage resides in non-volatile storage along with the log. Finally, high-performance database management systems usually blend the two preceding configurations by implementing a cache for cell storage in a volatile medium, and a potentially independent multilevel memory management algorithm moves data between the cache and non-volatile cell storage.

Recording everything twice adds one significant complication to all-or-nothing atomicity because the system can crash between the time a change is logged and the time it is installed. To maintain all-or-nothing atomicity, logging systems follow a protocol that has two fundamental requirements. The first requirement is a constraint on the order of logging and installing. The second requirement is to run an explicit *recovery* procedure after every crash. (We saw a preview of the strategy of using a recovery procedure in Figure 9.7, which used a recovery procedure named CHECK_AND_REPAIR.)

9.3.2 Logging Protocols

There are several kinds of atomicity logs that vary in the order in which things are done and in the details of information logged. However, all of them involve the ordering constraint implied by the numbering of the arrows in Figure 9.17. The constraint is a version of the *golden rule of atomicity* (never modify the only copy), known as the *write-ahead-log* (WAL) protocol:

Write-ahead-log protocol

Log the update *before* installing it.

The reason is that logging appends but installing overwrites. If an application violates this protocol by installing an update before logging it and then for some reason must abort, or the system crashes, there is no systematic way to discover the installed update and, if necessary, reverse it. The write-ahead-log protocol ensures that if a crash occurs, a recovery procedure can, by consulting the log, systematically find all completed and intended changes to cell storage and either restore those records to old values or set them to new values, as appropriate to the circumstance.

The basic element of an atomicity log is the *log record*. Before an action that is to be all-or-nothing installs a data value, it appends to the end of the log a new record of type CHANGE containing, in the general case, three pieces of information (we will later see special cases that allow omitting item 2 or item 3):

1. The identity of the all-or-nothing action that is performing the update.
2. A component action that, if performed, installs the intended value in cell storage. This component action is a kind of an insurance policy in case the system crashes. If the all-or-nothing action commits, but then the system crashes before the action has a chance to perform the install, the recovery procedure can perform the install

*database*在第一种常见配置中，单元存储完全位于某种易失性存储介质中。第二种常见配置下，单元存储与日志一同驻留在非易失性存储器内。最后，高性能数据库管理系统通常融合上述两种配置：通过在易失性介质中实现单元存储缓存，并由一个可能独立的多级内存管理算法在缓存与非易失性单元存储之间迁移数据。

记录所有内容两次为全有或全无原子性增添了一个重大复杂性，因为系统可能在更改被记录和安装之间崩溃。为了维持全有或全无原子性，日志系统遵循一个包含两项基本要求的协议。第一项要求是对记录和安装顺序的约束。第二项要求是在每次崩溃后运行一个明确的*recovery*过程。（我们在图9.7中预览了使用恢复过程的策略，该图使用了一个名为CHECK_AND_REPAIR的恢复过程。）

9.3.2 日志协议

存在多种原子性日志，它们在操作顺序和记录信息的细节上各有不同。然而，所有这些日志都涉及图9.17中箭头编号所隐含的顺序约束。该约束是*golden rule of atomicity*（永不修改唯一副本）的一个变体，被称为*write-ahead-log*（预写日志，~~WAL）协议~~：

预写式日志协议

记录更新 ***before*** 的安装过程。

原因在于日志记录是追加操作，而安装则是覆盖操作。如果某个应用违反此协议，在记录更新前就安装更新，随后又因故必须中止，或系统崩溃，就没有系统化的方法来发现已安装的更新，并在必要时撤销它。预写式日志协议确保一旦发生崩溃，恢复程序能通过查阅日志，系统性地找出所有已完成及预期的存储单元变更，并根据实际情况将这些记录恢复为旧值或更新为新值。

原子性日志的基本元素是*log record*。在执行一个全有或全无操作之前，当它要安装一个数据值时，会在日志末尾追加一条类型为CHANGE的新记录。在一般情况下，该记录包含三部分信息（后续我们将看到允许省略第2项或第3项的特殊情况）：

1. 执行更新的全有或全无操作的身份。
2. 一个组件操作，若执行，则会在单元格存储中安装预期值。此组件操作相当于一种保险措施，以防系统崩溃。如果全有或全无操作已提交，但系统在有机会执行安装前崩溃，恢复程序可以执行该安装。

on behalf of the action. Some systems call this component action the *do* action, others the *redo* action. For mnemonic compatibility with item 3, this text calls it the redo action.

3. A second component action that, if performed, reverses the effect on cell storage of the planned install. This component action is known as the *undo* action because if, after doing the install, the all-or-nothing action aborts or the system crashes, it may be necessary for the recovery procedure to reverse the effect of (*undo*) the install.

An application appends a log record by invoking the lower-layer procedure LOG, which itself must be atomic. The LOG procedure is another example of bootstrapping: Starting with, for example, the ALL_OR_NOTHING_PUT described earlier in this chapter, a log designer creates a generic LOG procedure, and using the LOG procedure an application programmer then can implement all-or-nothing atomicity for any properly designed composite action.

As we saw in Figure 9.17, LOG and INSTALL are the logging implementation of the WRITE_NEW_VALUE part of the interface of Figure 9.11, and READ_CURRENT_VALUE is simply a READ from cell storage. We also need a logging implementation of the remaining parts of the Figure 9.11 interface. The way to implement NEW_ACTION is to log a BEGIN record that contains just the new all-or-nothing action's identity. As the all-or-nothing action proceeds through its pre-commit phase, it logs CHANGE records. To implement COMMIT or ABORT, the all-or-nothing action logs an OUTCOME record that becomes the authoritative indication of the outcome of the all-or-nothing action. The instant that the all-or-nothing action logs the OUTCOME record is its commit point. As an example, Figure 9.19 shows our by now familiar TRANSFER action implemented with logging.

Because the log is the authoritative record of the action, the all-or-nothing action can perform installs to cell storage at any convenient time that is consistent with the write-ahead-log protocol, either before or after logging the OUTCOME record. The final step of an action is to log an END record, again containing just the action's identity, to show that the action has completed all of its installs. (Logging all four kinds of activity—BEGIN, CHANGE, OUTCOME, and END—is more general than sometimes necessary. As we will see, some logging systems can combine, e.g., OUTCOME and END, or BEGIN with the first CHANGE.) Figure 9.20 shows examples of three log records, two typical CHANGE records of an all-or-nothing TRANSFER action, interleaved with the OUTCOME record of some other, perhaps completely unrelated, all-or-nothing action.

One consequence of installing results in cell storage is that for an all-or-nothing action to abort it may have to do some clean-up work. Moreover, if the system involuntarily terminates a thread that is in the middle of an all-or-nothing action (because, for example, the thread has gotten into a deadlock or an endless loop) some entity other than the hapless thread must clean things up. If this clean-up step were omitted, the all-or-nothing action could remain pending indefinitely. The system cannot simply ignore indefinitely pending actions because all-or-nothing actions initiated by other threads are likely to want to use the data that the terminated action changed. (This is actually a

代表该动作。有些系统称这一动作为*do*动作，其他系统则称之为*redo*动作。为了与第3项的助记兼容性，本文将其称为重做动作。

3. 第二个组件动作，如果执行，将逆转计划安装对单元存储的影响。该组件动作被称为*undo*动作，因为如果在安装后，全有或全无动作中止或系统崩溃，恢复过程可能需要逆转(*undo*)安装的影响。

应用程序通过调用下层过程LOG追加日志记录，该过程本身必须是原子性的。LOG过程是另一个引导示例：从本章前文所述的ALL_OR_NOTHING_PUT出发，日志设计者创建了一个通用的LOG 过程，应用程序开发者随后可利用LOG 过程，为任何设计合理的复合动作实现全有或全无的原子性。

如图9.17所示，LOG和INSTALL实现了图9.11接口中WRITE_NEW_VALUE部分的日志记录功能，而READ_CURRENT_VALUE则直接来自单元存储的READ。我们还需要为图9.11接口的其余部分实现日志记录功能。实现NEW_ACTION的方法是记录一个仅包含新全有或全无操作标识的BEGIN记录。当全有或全无操作进入预提交阶段时，它会记录CHANGE记录。要实现COMMIT或ABORT，全有或全无操作会记录一个OUTCOME记录，该记录将成为操作结果的权威指示。全有或全无操作记录OUTCOME记录的瞬间即为提交点。例如，图9.19展示了我们现已熟悉的采用日志机制实现的TRANSFER操作。

由于日志是操作行为的权威记录，全有或全无操作可以在任何符合预写日志协议（write-ahead-log protocol）的适当时机，向单元存储执行安装操作，这一过程既可在记录OUTCOME之前，也可在其后进行。操作的最终步骤是记录一个END 条目，该条目同样仅包含操作标识，用以表明该操作已完成所有安装任务。（记录全部四种活动类型——BEGIN, CHANGE、OUTCOME, 和END———虽比实际需求更为通用。如后文所述，某些日志系统能够合并记录，例如将OUTCOME与END合并，或将BEGIN与首个CHANGE合并。）图9.20展示了三条日志记录的示例，其中两条是典型全有或全无TRANSFER操作的CHANGE记录，它们与另一可能完全无关的全有或全无操作的OUTCOME记录交错排列。

将结果安装到单元存储的一个后果是，对于全有或全无操作而言，若要中止它，可能需要进行一些清理工作。此外，如果系统非自愿地终止了一个正处于全有或全无操作中的线程（例如，因为该线程陷入了死锁或无限循环），那么必须由该不幸线程之外的某个实体来负责清理工作。如果省略了这一清理步骤，全有或全无操作可能会无限期地保持挂起状态。系统不能简单地无限期忽略这些挂起的操作，因为由其他线程发起的全有或全无操作很可能需要使用被终止操作所更改的数据。（这实际上是一个{v*}

```
1   procedure TRANSFER (debit_account, credit_account, amount)
2       my_id ← LOG (BEGIN_TRANSACTION)
3       dbvalue.old ← GET (debit_account)
4       dbvalue.new ← dbvalue.old - amount
5       crvalue.old ← GET (credit_account, my_id)
6       crvalue.new ← crvalue.old + amount
7       LOG (CHANGE, my_id,
8           "PUT (debit_account, dbvalue.new)",              //redo action
9           "PUT (debit_account, dbvalue.old)" )             //undo action
10      LOG ( CHANGE, my_id,
11          "PUT (credit_account, crvalue.new)"              //redo action
12          "PUT (credit_account, crvalue.old)")             //undo action
13      PUT (debit_account, dbvalue.new)                     // install
14      PUT (credit_account, crvalue.new)                    // install
15      if dbvalue.new > 0 then
16          LOG ( OUTCOME, COMMIT, my_id)
17      else
18          LOG (OUTCOME, ABORT, my_id)
19          signal("Action not allowed. Would make debit account negative.")
20      LOG (END_TRANSACTION, my_id)
```

FIGURE 9.19

An all-or-nothing TRANSFER procedure, implemented with logging.

before-or-after atomicity concern, one of the places where all-or-nothing atomicity and before-or-after atomicity intersect.)

If the action being aborted did any installs, those installs are still in cell storage, so simply appending to the log an OUTCOME record saying that the action aborted is not enough to make it appear to later observers that the all-or-nothing action did nothing. The solution to this problem is to execute a generic ABORT procedure. The ABORT proce-

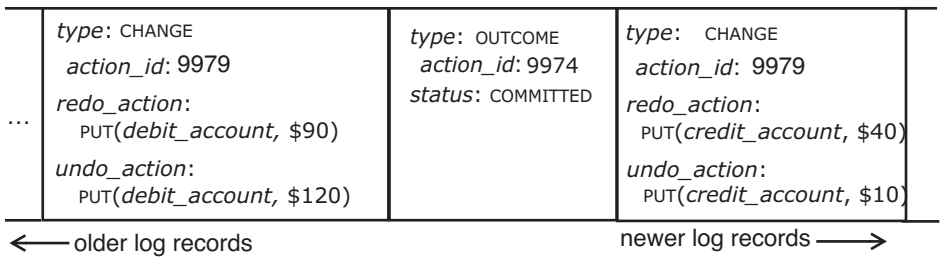

FIGURE 9.20

An example of a section of an atomicity log, showing two CHANGE records for a TRANSFER action that has action_id 9979 and the OUTCOME record of a different all-or-nothing action.

```
1   procedure TRANSFER (debit_account, credit_account, amount)
2       my_id ← LOG (BEGIN_TRANSACTION)
3       dbvalue.old ← GET (debit_account)
4       dbvalue.new ← dbvalue.old - amount
5       crvalue.old ← GET (credit_account, my_id)
6       crvalue.new ← crvalue.old + amount
7       LOG (CHANGE, my_id,
8           "PUT (debit_account, dbvalue.new)",              //redo action
9           "PUT (debit_account, dbvalue.old)" )             //undo action
10      LOG ( CHANGE, my_id,
11          "PUT (credit_account, crvalue.new)"              //redo action
12          "PUT (credit_account, crvalue.old)")             //undo action
13      PUT (debit_account, dbvalue.new)                     // install
14      PUT (credit_account, crvalue.new)                    // install
15      if dbvalue.new > 0 then
16          LOG ( OUTCOME, COMMIT, my_id)
17      else
18          LOG (OUTCOME, ABORT, my_id)
19          signal("Action not allowed. Would make debit account negative.")
20      LOG (END_TRANSACTION, my_id)
```

图9.19

一种全有或全无的TRANSFER程序，通过日志记录实现。

前后原子性关注点（这是全有或全无原子性与前后原子性相交的地方之一。）

如果被中止的操作执行了任何安装，这些安装仍保留在单元存储中，因此仅仅在日志中追加一条OUTCOME记录，声明操作已中止，并不足以让后续观察者认为这个全有或全无的操作什么都没做。解决这一问题的办法是执行一个通用的ABORT过程。该ABORT过程

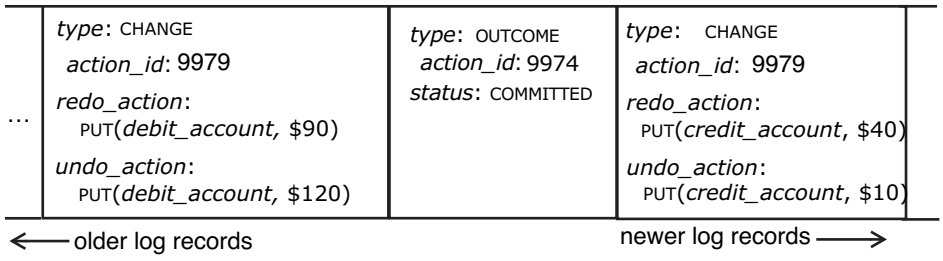

图9.20

原子性日志的一个片段示例，展示了针对动作ID为9979的TRANSFER操作的两条CHANGE记录，以及另一不同全有或全无动作的OUTCOME记录。

dure restores to their old values all cell storage variables that the all-or-nothing action installed. The ABORT procedure simply scans the log backwards looking for log entries created by this all-or-nothing action; for each CHANGE record it finds, it performs the logged *undo_action*, thus restoring the old values in cell storage. The backward search terminates when the ABORT procedure finds that all-or-nothing action's BEGIN record. Figure 9.21 illustrates.

The extra work required to undo cell storage installs when an all-or-nothing action aborts is another example of *optimizing for the common case*: one expects that most all-or-nothing actions will commit, and that aborted actions should be relatively rare. The extra effort of an occasional roll back of cell storage values will (one hopes) be more than repaid by the more frequent gains in performance on updates, reads, and commits.

9.3.3 Recovery Procedures

The write-ahead log protocol is the first of the two required protocol elements of a logging system. The second required protocol element is that, following every system crash, the system must run a recovery procedure before it allows ordinary applications to use the data. The details of the recovery procedure depend on the particular configuration of the journal and cell storage with respect to volatile and non-volatile memory.

Consider first recovery for the in-memory database of Figure 9.18. Since a system crash may corrupt anything that is in volatile memory, including both the state of cell storage and the state of any currently running threads, restarting a crashed system usually begins by resetting all volatile memory. The effect of this reset is to abandon both the cell

```
1  procedure ABORT (action_id)
2      starting at end of log repeat until beginning
3          log_record ← previous record of log
4          if log_record.id = action_id then
5              if (log_record.type = OUTCOME)
6                  then signal ("Can't abort an already completed action.")
7              if (log_record.type = CHANGE)
8                  then perform undo_action of log_record
9              if (log_record.type = BEGIN)
10                 then break repeat
11     LOG (action_id, OUTCOME, ABORTED)                    // Block future undos.
12     LOG (action_id, END)
```

FIGURE 9.21

Generic ABORT procedure for a logging system. The argument *action_id* identifies the action to be aborted. An atomic action calls this procedure if it decides to abort. In addition, the operating system may call this procedure if it decides to terminate the action, for example to break a deadlock or because the action is running too long. The LOG procedure must itself be atomic.

dure将所有由全有或全无操作设置的单元格存储变量恢复至其旧值。ABORT过程简单地逆向扫描日志，寻找由该全有或全无操作创建的日志条目；对于发现的每一个CHANGE记录，它执行记录的*undo_action*,操作，从而恢复单元格存储中的旧值。当ABORT过程找到该全有或全无操作的BEGIN记录时，逆向搜索终止。图9.21对此进行了说明。

撤销单元格存储安装所需的额外工作，当全有或全无操作中止时，是*optimizing for the common case*的另一个例证：人们预期大多数全有或全无操作将会提交，而中止操作应相对罕见。偶尔回滚单元格存储值的额外努力（但愿如此）将通过更频繁地在更新、读取和提交操作上获得的性能提升得到超额补偿。

9.3.3 恢复流程

预写式日志协议是日志系统中两个必需协议要素中的第一个。第二个必需协议要素是，在每次系统崩溃后，系统必须运行恢复程序，才能允许普通应用程序使用数据。恢复程序的具体细节取决于日志与单元存储针对易失性和非易失性内存的特定配置。

首先考虑图9.18中内存数据库的恢复。由于系统崩溃可能损坏易失性内存中的任何内容，包括单元存储状态和当前运行线程的状态，重启崩溃系统通常从重置所有易失性内存开始。这种重置的效果是放弃所有单元

```
1  procedure ABORT (action_id)
2      starting at end of log repeat until beginning
3          log_record ← previous record of log
4          if log_record.id = action_id then
5              if (log_record.type = OUTCOME)
6                  then signal ("Can't abort an already completed action.")
7              if (log_record.type = CHANGE)
8                  then perform undo_action of log_record
9              if (log_record.type = BEGIN)
10                 then break repeat
11     LOG (action_id, OUTCOME, ABORTED)                    // Block future undos.
12     LOG (action_id, END)
```

图9.21

Generic ABORT procedure for a logging system. The argument *action_id* identifies the action to be aborted. An atomic action calls this procedure if it decides to abort. In addition, the operating system may call this procedure if it decides to terminate the action, for example to break a deadlock or because the action is running too long. The LOG procedure must itself be atomic.

```
1  procedure RECOVER () // Recovery procedure for a volatile, in-memory database.
2      winners ← NULL
3      starting at end of log repeat until beginning
4          log_record ← previous record of log
5          if (log_record.type = OUTCOME)
6              then winners ← winners + log_record                    // Set addition.

7      starting at beginning of log repeat until end
8          log_record ← next record of log
9          if (log_record.type= CHANGE)
10             and (outcome_record ← find (log_record.action_id) in winners)
11             and (outcome_record.status = COMMITTED) then
12             perform log_record.redo_action
```

FIGURE 9.22

An idempotent redo-only recovery procedure for an in-memory database. Because RECOVER writes only to volatile storage, if a crash occurs while it is running it is safe to run it again.

storage version of the database and any all-or-nothing actions that were in progress at the time of the crash. On the other hand, the log, since it resides on non-volatile journal storage, is unaffected by the crash and should still be intact.

The simplest recovery procedure performs two passes through the log. On the first pass, it scans the log *backward* from the last record, so the first evidence it will encounter of each all-or-nothing action is the last record that the all-or-nothing action logged. A backward log scan is sometimes called a LIFO (for last-in, first-out) log review. As the recovery procedure scans backward, it collects in a set the identity and completion status of every all-or-nothing action that logged an OUTCOME record before the crash. These actions, whether committed or aborted, are known as *winners*.

When the backward scan is complete the set of winners is also complete, and the recovery procedure begins a forward scan of the log. The reason the forward scan is needed is that restarting after the crash completely reset the cell storage. During the forward scan the recovery procedure performs, in the order found in the log, all of the REDO actions of every winner whose OUTCOME record says that it COMMITTED. Those REDOs reinstall all committed values in cell storage, so at the end of this scan, the recovery procedure has restored cell storage to a desirable state. This state is as if every all-or-nothing action that committed before the crash had run to completion, while every all-or-nothing action that aborted or that was still pending at crash time had never existed. The database system can now open for regular business. Figure 9.22 illustrates.

This recovery procedure emphasizes the point that a log can be viewed as an authoritative version of the entire database, sufficient to completely reconstruct the reference copy in cell storage.

There exist cases for which this recovery procedure may be overkill, when the durability requirement of the data is minimal. For example, the all-or-nothing action may

```
1  procedure RECOVER () // Recovery procedure for a volatile, in-memory database.
2      winners ← NULL
3      starting at end of log repeat until beginning
4          log_record ← previous record of log
5          if (log_record.type = OUTCOME)
6              then winners ← winners + log_record                    // Set addition.

7      starting at beginning of log repeat until end
8          log_record ← next record of log
9          if (log_record.type= CHANGE)
10             and (outcome_record ← find (log_record.action_id) in winners)
11             and (outcome_record.status = COMMITTED) then
12             perform log_record.redo_action
```

图9.22

一种针对内存数据库的幂等仅重做恢复过程。由于RECOVER 仅写入易失性存储，如果在运行过程中发生崩溃，再次运行它是安全的。

数据库的存储版本以及崩溃时正在进行的所有全有或全无操作。另一方面，由于日志驻留在非易失性日志存储上，它不受崩溃影响，应仍保持完整。

最简单的恢复过程会对日志进行两次扫描。第一次扫描时，它从最后一条记录开始逆向检查日志*backward*，因此对于每个全有或全无操作，恢复过程首先遇到的证据将是该操作最后记录的日志条目。这种逆向日志扫描有时被称为LIFO（后进先出）式日志审查。在逆向扫描过程中，恢复程序会收集所有在崩溃前记录过OUTCOME条目的全有或全无操作的身份标识和完成状态，存入一个集合中。这些操作，无论已提交还是已中止，都被称为*winners*。

当后向扫描完成时，获胜者集合也随之确定，恢复过程随即开始对日志进行前向扫描。之所以需要前向扫描，是因为崩溃后的重启完全重置了单元存储。在前向扫描过程中，恢复程序按照日志中的顺序，执行每个获胜者的REDO 动作——只要其OUTCOME记录表明该动作COMMITTED。这些REDO操作会将所有已提交的值重新载入单元存储，因此扫描结束时，恢复程序已将单元存储恢复到理想状态。这一状态如同所有在崩溃前提交的全有或全无动作均已完成，而所有中止或在崩溃时仍处于待定状态的全有或全无动作从未存在过。此时数据库系统便可重新开放进行正常操作。图9.22对此进行了说明。

这一恢复流程强调了一个观点：日志可被视为整个数据库的权威版本，足以完整重构单元存储中的参考副本{v*}。

在某些情况下，当数据的持久性要求极低时，这种恢复流程可能显得过于繁琐。例如，全有或全无的操作可能

have been to make a group of changes to soft state in volatile storage. If the soft state is completely lost in a crash, there would be no need to redo installs because the definition of soft state is that the application is prepared to construct new soft state following a crash. Put another way, given the options of "all" or "nothing," when the data is all soft state "nothing" is always an appropriate outcome after a crash.

A critical design property of the recovery procedure is that, if there should be another system crash during recovery, it must still be possible to recover. Moreover, it must be possible for any number of crash-restart cycles to occur without compromising the correctness of the ultimate result. The method is to design the recovery procedure to be *idempotent*. That is, design it so that if it is interrupted and restarted from the beginning it will produce exactly the same result as if it had run to completion to begin with. With the in-memory database configuration, this goal is an easy one: just make sure that the recovery procedure modifies only volatile storage. Then, if a crash occurs during recovery, the loss of volatile storage automatically restores the state of the system to the way it was when the recovery started, and it is safe to run it again from the beginning. If the recovery procedure ever finishes, the state of the cell storage copy of the database will be correct, no matter how many interruptions and restarts intervened.

The ABORT procedure similarly needs to be idempotent because if an all-or-nothing action decides to abort and, while running ABORT, some timer expires, the system may decide to terminate and call ABORT for that same all-or-nothing action. The version of abort in Figure 9.21 will satisfy this requirement if the individual undo actions are themselves idempotent.

9.3.4 Other Logging Configurations: Non-Volatile Cell Storage

Placing cell storage in volatile memory is a *sweeping simplification* that works well for small and medium-sized databases, but some databases are too large for that to be practical, so the designer finds it necessary to place cell storage on some cheaper, non-volatile storage medium such as magnetic disk, as in the second configuration of Figure 9.18. But with a non-volatile storage medium, installs survive system crashes, so the simple recovery procedure used with the in-memory database would have two shortcomings:

1. If, at the time of the crash, there were some pending all-or-nothing actions that had installed changes, those changes will survive the system crash. The recovery procedure must reverse the effects of those changes, just as if those actions had aborted.
2. That recovery procedure reinstalls the entire database, even though in this case much of it is probably intact in non-volatile storage. If the database is large enough that it requires non-volatile storage to contain it, the cost of unnecessarily reinstalling it in its entirety at every recovery is likely to be unacceptable.

In addition, reads and writes to non-volatile cell storage are likely to be slow, so it is nearly always the case that the designer installs a cache in volatile memory, along with a

对易失性存储中的软状态进行了一系列更改。如果崩溃导致软状态完全丢失，也无需重新执行安装操作，因为软状态的定义就是应用程序能够在崩溃后重建新的软状态。换句话说，在“全有”或“全无”的选择中，当数据全是软状态时，崩溃后“全无”始终是一个合适的结果。

恢复过程的一个关键设计特性是，即便在恢复期间系统再次崩溃，也必须能够继续恢复。此外，必须确保无论发生多少次崩溃-重启循环，都不会影响最终结果的正确性。其方法是将恢复过程设计为*idempotent*。也就是说，设计时要确保即使恢复过程被中断并从开头重新启动，其产生的结果与完整运行至结束时的结果完全相同。对于内存数据库配置而言，这一目标很容易实现：只需确保恢复过程仅修改易失性存储。这样，若恢复过程中发生崩溃，易失性存储的丢失会自动将系统状态恢复到恢复开始时的状态，从而可以安全地从头重新运行恢复过程。只要恢复过程最终完成，无论中间经历了多少次中断和重启，数据库的单元存储副本状态都将保持正确。

ABORT 过程同样需要是幂等的，因为如果一个全有或全无操作决定中止，并且在执行ABORT过程中某些计时器到期，系统可能会决定终止并为同一个全有或全无操作调用ABORT。如果各个撤销操作本身是幂等的，那么图9.21中的中止版本将满足这一要求。

9.3.4 其他日志配置：非易失性单元存储

将单元格存储置于易失性内存是一种*sweeping simplification*，对于中小型数据库运行良好，但某些数据库规模过大，这种做法便不切实际。因此，设计者发现有必要将单元格存储置于更廉价、非易失性的存储介质上，例如图9.18第二种配置中的磁盘。然而，非易失性存储介质使得安装操作能在系统崩溃后保留，这样一来，内存数据库所采用的简单恢复程序便暴露出两大缺陷：

1. 如果在系统崩溃时，存在一些已安装更改的“全有或全无”待处理操作，这些更改将在系统崩溃后保留。恢复过程必须撤销这些更改的影响，就像这些操作已中止一样。 2. 该恢复过程会重新安装整个数据库，尽管在此情况下，大部分数据可能在非易失性存储中完好无损。如果数据库规模庞大，需要使用非易失性存储来容纳，那么每次恢复时都毫无必要地重新完整安装整个数据库，其成本很可能是无法接受的。

此外，对非易失性单元存储器的读写操作往往较慢，因此设计者几乎总是会在易失性存储器中安装一个缓存，并配备一个

multilevel memory manager, thus moving to the third configuration of Figure 9.18. But that addition introduces yet another shortcoming:

3. In a multilevel memory system, the order in which data is written from volatile levels to non-volatile levels is generally under control of a multilevel memory manager, which may, for example, be running a least-recently-used algorithm. As a result, at the instant of the crash some things that were thought to have been installed may not yet have migrated to the non-volatile memory.

To postpone consideration of this shortcoming, let us for the moment assume that the multilevel memory manager implements a write-through cache. (Section 9.3.6, below, will return to the case where the cache is not write-through.) With a write-through cache, we can be certain that everything that the application program has installed has been written to non-volatile storage. This assumption temporarily drops the third shortcoming out of our list of concerns and the situation is the same as if we were using the "Ordinary Database" configuration of Figure 9.18 with no cache. But we still have to do something about the first two shortcomings, and we also must make sure that the modified recovery procedure is still idempotent.

To address the first shortcoming, that the database may contain installs from actions that should be undone, we need to modify the recovery procedure of Figure 9.22. As the recovery procedure performs its initial backward scan, rather than looking for winners, it instead collects in a set the identity of those all-or-nothing actions that were still in progress at the time of the crash. The actions in this set are known as *losers*, and they can include both actions that committed and actions that did not. Losers are easy to identify because the first log record that contains their identity that is encountered in a backward scan will be something other than an END record. To identify the losers, the pseudocode keeps track of which actions logged an END record in an auxiliary list named *completeds*. When RECOVER comes across a log record belong to an action that is not in *completed*, it adds that action to the set named *losers*. In addition, as it scans backwards, whenever the recovery procedure encounters a CHANGE record belonging to a loser, it performs the UNDO action listed in the record. In the course of the LIFO log review, all of the installs performed by losers will thus be rolled back and the state of the cell storage will be as if the all-or-nothing actions of losers had never started. Next, RECOVER performs the forward log scan of the log, performing the redo actions of the all-or-nothing actions that committed, as shown in Figure 9.23. Finally, the recovery procedure logs an END record for every all-or-nothing action in the list of losers. This END record transforms the loser into a completed action, thus ensuring that future recoveries will ignore it and not perform its undos again. For future recoveries to ignore aborted losers is not just a performance enhancement, it is essential, to avoid incorrectly undoing updates to those same variables made by future all-or-nothing actions.

As before, the recovery procedure must be idempotent, so that if a crash occurs during recovery the system can just run the recovery procedure again. In addition to the technique used earlier of placing the temporary variables of the recovery procedure in volatile storage, each individual undo action must also be idempotent. For this reason, both redo

多级内存管理器，从而转向图9.18的第三种配置。但这一新增又引入了另一个缺点：

3. 在多级存储系统中，数据从易失性层级写入非易失性层级的顺序通常由多级存储管理器控制，该管理器可能运行诸如最近最少使用算法等策略。因此，在系统崩溃的瞬间，一些被认为已安装的数据可能尚未迁移至非易失性存储器中。

为了暂缓考虑这一缺陷，让我们暂且假设多级内存管理器实现了写透缓存（write-through cache）。（下文第9.3.6节将重新讨论非写透缓存的情况。）采用写透缓存时，我们可以确信应用程序已安装的所有内容都已写入非易失性存储。这一假设暂时将第三个缺陷从我们关注的问题列表中移除，此时的情况就如同我们使用图9.18中无缓存的“普通数据库”配置一样。但我们仍需解决前两个缺陷，同时必须确保修改后的恢复过程仍保持幂等性。

为了解决第一个缺陷，即数据库中可能包含需要撤销操作的安装记录，我们需要修改图9.22中的恢复流程。在恢复过程执行初始反向扫描时，不再寻找成功操作，而是将所有在崩溃时仍在进行中的全有或全无操作的身份信息收集到一个集合中。这个集合中的操作被称为*losers*，它们可能包括已提交和未提交的操作。失败操作很容易识别，因为在反向扫描中遇到的第一个包含其身份信息的日志记录将不是END记录。为了识别失败操作，伪代码在一个名为*completeds*的辅助列表中跟踪哪些操作记录了END记录。当RECOVER遇到一个不属于*completed*的操作的日志记录时，它会将该操作添加到名为*losers*的集合中。此外，在反向扫描过程中，每当恢复过程遇到属于失败操作的CHANGE记录时，它会执行记录中列出的UNDO 操作。通过这种后进先出的日志审查，所有由失败操作执行的安装都将被回滚，单元格存储的状态将如同这些失败的全有或全无操作从未开始过一样。接下来，RECOVER执行日志的正向扫描，对已提交的全有或全无操作执行重做操作，如图9.23所示。最后，恢复过程为失败操作列表中的每一个全有或全无操作记录一个END记录。这个END记录将失败操作转化为已完成操作，从而确保未来的恢复过程会忽略它，不会再次执行其撤销操作。让未来的恢复过程忽略已中止的失败操作不仅是一种性能优化，更是为了避免错误地撤销未来全有或全无操作对相同变量所做的更新，这一点至关重要。

与之前一样，恢复过程必须是幂等的，这样如果在恢复过程中发生崩溃，系统可以再次运行恢复程序。除了之前使用的将恢复过程的临时变量置于易失性存储的技术外，每个单独的撤销操作也必须是幂等的。因此，无论是重做

```
1  procedure RECOVER ()// Recovery procedure for non-volatile cell memory
2      completeds ← NULL
3      losers ← NULL
4      starting at end of log repeat until beginning
5          log_record ← previous record of log
6          if (log_record.type = END)
7              then completeds ← completeds + log_record          // Set addition.
8          if (log_record.action_id is not in completeds) then
9              losers ← losers + log_record                  // Add if not already in set.
10             if (log_record.type = CHANGE) then
11                 perform log_record.undo_action

12     starting at beginning of log repeat until end
13         log_record ← next record of log
14         if (log_record.type = CHANGE)
15             and (log_record.action_id.status = COMMITTED) then
16             perform log_record.redo_action

17     for each log_record in losers do
18         log (log_record.action_id, END)                    // Show action completed.
```

FIGURE 9.23

An idempotent undo/redo recovery procedure for a system that performs installs to non-volatile cell memory. In this recovery procedure, *losers* are all-or-nothing actions that were in progress at the time of the crash.

and undo actions are usually expressed as *blind writes*. A blind write is a simple overwriting of a data value without reference to its previous value. Because a blind write is inherently idempotent, no matter how many times one repeats it, the result is always the same. Thus, if a crash occurs part way through the logging of END records of losers, immediately rerunning the recovery procedure will still leave the database correct. Any losers that now have END records will be treated as completed on the rerun, but that is OK because the previous attempt of the recovery procedure has already undone their installs.

As for the second shortcoming, that the recovery procedure unnecessarily redoes every install, even installs not belong to losers, we can significantly simplify (and speed up) recovery by analyzing why we have to redo any installs at all. The reason is that, although the WAL protocol requires logging of changes to occur before install, there is no necessary ordering between commit and install. Until a committed action logs its END record, there is no assurance that any particular install of that action has actually happened yet. On the other hand, any committed action that has logged an END record has completed its installs. The conclusion is that the recovery procedure does not need to

```
1  procedure RECOVER ()// Recovery procedure for non-volatile cell memory
2      completeds ← NULL
3      losers ← NULL
4      starting at end of log repeat until beginning
5          log_record ← previous record of log
6          if (log_record.type = END)
7              then completeds ← completeds + log_record          // Set addition.
8          if (log_record.action_id is not in completeds) then
9              losers ← losers + log_record                 // Add if not already in set.
10             if (log_record.type = CHANGE) then
11                 perform log_record.undo_action

12     starting at beginning of log repeat until end
13         log_record ← next record of log
14         if (log_record.type = CHANGE)
15             and (log_record.action_id.status = COMMITTED) then
16             perform log_record.redo_action

17     for each log_record in losers do
18         log (log_record.action_id, END)                    // Show action completed.
```

图9.23

一种针对执行非易失性单元内存安装系统的幂等撤销/重做恢复流程。在此恢复流程中，*losers* 代表系统崩溃时正在进行的所有或全无操作。

撤销操作通常表示为*blind writes*。盲写是指简单地覆盖数据值，而不参考其先前值。由于盲写本身具有幂等性，无论重复执行多少次，结果始终相同。因此，如果在记录失败者的END记录过程中发生崩溃，立即重新运行恢复程序仍能保持数据库的正确性。那些现在拥有END记录的失败者在重新运行时将被视为已完成，但这没有问题，因为恢复程序的上一次尝试已经撤销了它们的安装。

至于第二个缺点，即恢复过程不必要地重做每一个安装操作，甚至包括那些不属于失败者的安装，我们可以通过分析为何需要重做任何安装来大幅简化（并加速）恢复过程。原因在于，尽管WAL协议要求在安装前记录变更，但提交与安装之间并无必然的顺序要求。在某个已提交的操作记录其END 记录之前，无法确保该操作的任何特定安装已经实际发生。另一方面，任何已记录END记录的已提交操作，其安装过程已经完成。由此得出的结论是，恢复过程无需

```
1  procedure RECOVER ()              // Recovery procedure for rollback recovery.
2      completeds ← NULL
3      losers ← NULL
4      starting at end of log repeat until beginning          // Perform undo scan.
5          log_record ← previous record of log
6          if (log_record.type = OUTCOME)
7              then completeds ← completeds + log_record      // Set addition.
8          if (log_record.action_id is not in completeds) then
9              losers ← losers + log_record                   // New loser.
10             if (log_record.type = CHANGE) then
11                 perform log_record.undo_action

12     for each log_record in losers do
13         log (log_record.action_id, OUTCOME, ABORT)         // Block future undos.
```

FIGURE 9.24

An idempotent undo-only recovery procedure for rollback logging.

redo installs for any committed action that has logged its END record. A useful exercise is to modify the procedure of Figure 9.23 to take advantage of that observation.

It would be even better if the recovery procedure never had to redo *any* installs. We can arrange for that by placing another requirement on the application: it must perform all of its installs *before* it logs its OUTCOME record. That requirement, together with the write-through cache, ensures that the installs of every completed all-or-nothing action are safely in non-volatile cell storage and there is thus never a need to perform *any* redo actions. (It also means that there is no need to log an END record.) The result is that the recovery procedure needs only to undo the installs of losers, and it can skip the entire forward scan, leading to the simpler recovery procedure of Figure 9.24. This scheme, because it requires only undos, is sometimes called *undo logging* or *rollback recovery*. A property of rollback recovery is that for completed actions, cell storage is just as authoritative as the log. As a result, one can garbage collect the log, discarding the log records of completed actions. The now much smaller log may then be able to fit in a faster storage medium for which the durability requirement is only that it outlast pending actions.

There is an alternative, symmetric constraint used by some logging systems. Rather than requiring that all installs be done *before* logging the OUTCOME record, one can instead require that all installs be done *after* recording the OUTCOME record. With this constraint, the set of CHANGE records in the log that belong to that all-or-nothing action become a description of its intentions. If there is a crash before logging an OUTCOME record, we know that no installs have happened, so the recovery never needs to perform any undos. On the other hand, it may have to perform installs for all-or-nothing actions that committed. This scheme is called *redo logging* or *roll-forward recovery*. Furthermore, because we are uncertain about which installs actually have taken place, the recovery procedure must

```
1  procedure RECOVER ()          // Recovery procedure for rollback recovery.
2      completeds ← NULL
3      losers ← NULL
4      starting at end of log repeat until beginning          // Perform undo scan.
5          log_record ← previous record of log
6          if (log_record.type = OUTCOME)
7              then completeds ← completeds + log_record      // Set addition.
8          if (log_record.action_id is not in completeds) then
9              losers ← losers + log_record                   // New loser.
10             if (log_record.type = CHANGE) then
11                 perform log_record.undo_action

12     for each log_record in losers do
13         log (log_record.action_id, OUTCOME, ABORT)         // Block future undos.
```

图9.24

一种幂等的仅撤销回滚日志恢复过程。

对于任何已记录其END记录且已提交的操作，redo会重新执行安装。一个有益的练习是修改图9.23中的流程，以利用这一观察结果。

如果恢复过程永远不需要重做*any*安装，那就更好了。我们可以通过在应用程序上添加另一个要求来实现这一点：它必须在记录其OUTCOME记录之前完成所有*before*安装。这一要求，加上直写缓存，确保了每个完成的“全有或全无”操作的安装都已安全地存储在非易失性单元存储器中，因此永远不需要执行*any*重做操作。（这也意味着无需记录END记录。）结果是恢复过程只需撤销失败者的安装，并且可以跳过整个前向扫描，从而得到图9.24中更简单的恢复过程。这种方案，因为它只需要撤销操作，有时被称为*undo logging*或*rollback recovery*。回滚恢复的一个特性是，对于已完成的操作，单元存储与日志具有同等的权威性。因此，可以垃圾回收日志，丢弃已完成操作的日志记录。现在小得多的日志可能能够放入更快的存储介质中，其持久性要求仅需超过待处理操作的寿命即可。

有一种替代的、对称性约束被某些日志系统所采用。不同于要求所有安装操作必须记录OUTCOME记录时完成*before*，可以转而要求所有安装操作在记录OUTCOME记录时完成*after*。在这种约束下，日志中属于该全有或全无动作的CHANGE记录集合便成为其意图的描述。如果在记录OUTCOME记录前发生崩溃，我们知道尚未执行任何安装操作，因此恢复过程无需执行任何撤销操作。另一方面，它可能需要为已提交的全有或全无动作执行安装操作。此方案被称为*redo logging*或*roll-forward recovery*。此外，由于我们无法确定哪些安装操作实际已发生，恢复过程必须

perform *all* logged installs for all-or-nothing actions that did not log an END record. Any all-or-nothing action that logged an END record must have completed all of its installs, so there is no need for the recovery procedure to perform them. The recovery procedure thus reduces to doing installs just for all-or-nothing actions that were interrupted between the logging of their OUTCOME and END records. Recovery with redo logging can thus be quite swift, though it does require both a backward and forward scan of the entire log.

We can summarize the procedures for atomicity logging as follows:

- Log to journal storage before installing in cell storage (WAL protocol)
- If all-or-nothing actions perform *all* installs to non-volatile storage before logging their OUTCOME record, then recovery needs only to undo the installs of incomplete uncommitted actions. (rollback/undo recovery)
- If all-or-nothing actions perform *no* installs to non-volatile storage before logging their OUTCOME record, then recovery needs only to redo the installs of incomplete committed actions. (roll-forward/redo recovery)
- If all-or-nothing actions are not disciplined about when they do installs to non-volatile storage, then recovery needs to both redo the installs of incomplete committed actions *and* undo the installs of incomplete uncommitted ones.

In addition to reading and updating memory, an all-or-nothing action may also need to send messages, for example, to report its success to the outside world. The action of sending a message is just like any other component action of the all-or-nothing action. To provide all-or-nothing atomicity, message sending can be handled in a way analogous to memory update. That is, log a CHANGE record with a redo action that sends the message. If a crash occurs after the all-or-nothing action commits, the recovery procedure will perform this redo action along with other redo actions that perform installs. In principle, one could also log an *undo_action* that sends a compensating message ("Please ignore my previous communication!"). However, an all-or-nothing action will usually be careful not to actually send any messages until after the action commits, so roll-forward recovery applies. For this reason, a designer would not normally specify an undo action for a message or for any other action that has outside-world visibility such as printing a receipt, opening a cash drawer, drilling a hole, or firing a missile.

Incidentally, although much of the professional literature about database atomicity and recovery uses the terms "winner" and "loser" to describe the recovery procedure, different recovery systems use subtly different definitions for the two sets, depending on the exact logging scheme, so it is a good idea to review those definitions carefully.

9.3.5 Checkpoints

Constraining the order of installs to be all before or all after the logging of the OUTCOME record is not the only thing we could do to speed up recovery. Another technique that can shorten the log scan is to occasionally write some additional information, known as a *checkpoint,* to non-volatile storage. Although the principle is always the same, the exact

对所有未记录END的全有或全无操作执行*all*的日志安装。任何记录了END的全有或全无操作必须已完成其所有安装，因此恢复过程无需再次执行。因此，恢复过程简化为仅对那些在记录OUTCOME和END之间被中断的全有或全无操作进行安装。尽管重做日志恢复需要前后扫描整个日志，但其过程可以相当迅速。

我们可以将原子性日志记录的程序总结如下:

- 在安装到单元存储之前记录到日志存储（WAL协议）
- 如果全有或全无操作在执行*all*安装到非易失性存储之前记录其OUTCOME记录，那么恢复时只需撤销未完成且未提交操作的安装。（回滚/撤销恢复）
- 如果全有或全无操作在记录其OUTCOME记录之前将*no*安装到非易失性存储中，那么恢复时只需重做未完成已提交操作的安装。（前滚/重做恢复）
- 如果全有或全无操作在何时安装到非易失性存储上缺乏纪律性，那么恢复过程不仅需要重做未完成已提交操作的安装*and*，还需撤销未完成未提交操作的安装。

除了读取和更新内存外，一个全有或全无操作可能还需要发送消息，例如向外界报告其成功。发送消息的行为就如同全有或全无操作中的任何其他组成部分一样。为了提供全有或全无的原子性，消息发送可以采用与内存更新类似的方式处理。即，记录一个CHANGE日志条目，其中包含一个用于发送消息的重做操作。如果在全有或全无操作提交后发生崩溃，恢复过程将执行此重做操作以及其他执行安装的重做操作。理论上，也可以记录一个*undo_action*日志条目来发送补偿消息（“请忽略我之前的信息！”）。然而，全有或全无操作通常会谨慎地在操作提交之前不实际发送任何消息，因此适用前滚恢复。基于此原因，设计者通常不会为消息或任何其他具有外界可见性的操作（如打印收据、打开现金抽屉、钻孔或发射导弹）指定撤销操作。

顺便提一下，尽管关于数据库原子性与恢复的大量专业文献采用“胜者”和“败者”这两个术语来描述恢复过程，但不同的恢复系统会依据具体的日志方案对这两组定义存在细微差异，因此仔细审阅这些定义是个明智的做法。

9.3.5 检查点

限制安装顺序必须在记录OUTCOME 之前或之后全部完成，并非我们加速恢复的唯一手段。另一种能缩短日志扫描的技术是偶尔向非易失性存储写入一些额外信息，即所谓的 *checkpoint*。尽管原理始终如一，但具体的

information that is placed in a checkpoint varies from one system to another. A checkpoint can include information written either to cell storage or to the log (where it is known as a *checkpoint record*) or both.

Suppose, for example, that the logging system maintains in volatile memory a list of identifiers of all-or-nothing actions that have started but have not yet recorded an END record, together with their pending/committed/aborted status, keeping it up to date by observing logging calls. The logging system then occasionally logs this list as a CHECKPOINT record. When a crash occurs sometime later, the recovery procedure begins a LIFO log scan as usual, collecting the sets of completed actions and losers. When it comes to a CHECKPOINT record it can immediately fill out the set of losers by adding those all-or-nothing actions that were listed in the checkpoint that did not later log an END record. This list may include some all-or-nothing actions listed in the CHECKPOINT record as COMMITTED, but that did not log an END record by the time of the crash. Their installs still need to be performed, so they need to be added to the set of losers. The LIFO scan continues, but only until it has found the BEGIN record of every loser.

With the addition of CHECKPOINT records, the recovery procedure becomes more complex, but is potentially shorter in time and effort:

1. Do a LIFO scan of the log back to the last CHECKPOINT record, collecting identifiers of losers and undoing all actions they logged.
2. Complete the list of losers from information in the checkpoint.
3. Continue the LIFO scan, undoing the actions of losers, until every BEGIN record belonging to every loser has been found.
4. Perform a forward scan from that point to the end of the log, performing any committed actions belonging to all-or-nothing actions in the list of losers that logged an OUTCOME record with status COMMITTED.

In systems in which long-running all-or-nothing actions are uncommon, step 3 will typically be quite brief or even empty, greatly shortening recovery. A good exercise is to modify the recovery program of Figure 9.23 to accommodate checkpoints.

Checkpoints are also used with in-memory databases, to provide durability without the need to reprocess the entire log after every system crash. A useful checkpoint procedure for an in-memory database is to make a snapshot of the complete database, writing it to one of two alternating (for all-or-nothing atomicity) dedicated non-volatile storage regions, and then logging a CHECKPOINT record that contains the address of the latest snapshot. Recovery then involves scanning the log back to the most recent CHECKPOINT record, collecting a list of committed all-or-nothing actions, restoring the snapshot described there, and then performing redo actions of those committed actions from the CHECKPOINT record to the end of the log. The main challenge in this scenario is dealing with update activity that is concurrent with the writing of the snapshot. That challenge can be met either by preventing all updates for the duration of the snapshot or by applying more complex before-or-after atomicity techniques such as those described in later sections of this chapter.

放置在检查点中的信息因系统而异。一个检查点可以包含写入单元存储的信息，或写入日志的信息（此时称为*checkpoint record*），或两者兼有。

例如，假设日志系统在易失性内存中维护一个列表，记录所有已启动但尚未记录END 记录的全有或全无操作的标识符及其挂起/已提交/已中止状态，并通过观察日志调用来保持该列表的最新状态。日志系统随后会偶尔将此列表作为CHECKPOINT 记录进行日志记录。当稍后发生崩溃时，恢复过程会像往常一样开始LIFO（后进先出）日志扫描，收集已完成操作和失败操作的集合。当遇到CHECKPOINT记录时，它可以通过添加那些在检查点中列出但后来未记录END记录的全有或全无操作，立即填充失败操作集合。此列表可能包含一些在CHECKPOINT记录中标记为COMMITTED的全有或全无操作，但在崩溃时尚未记录END记录。这些操作的安装仍需执行，因此需要将它们添加到失败操作集合中。LIFO扫描继续进行，但仅直到找到每个失败操作的BEGIN记录为止。

随着CHECKPOINT记录的加入，恢复过程变得更加复杂，但在时间和精力上可能更短：

1. 对日志进行LIFO（后进先出）扫描，回溯至最近的CHECKPOINT记录，收集失败者的标识并撤销其记录的所有操作。 2. 根据检查点中的信息完善失败者列表。 3. 继续LIFO扫描，撤销失败者的操作，直至找到每个失败者相关的所有BEGIN记录。 4. 从该点开始向前扫描至日志末尾，执行失败者列表中所有全有或全无动作已提交的操作，这些操作需记录有状态为COMMITTED的OUTCOME记录。

在那些长时间运行的“全有或全无”操作并不常见的系统中，第三步通常会非常简短，甚至为空，从而大大缩短恢复时间。一个很好的练习是修改图9.23中的恢复程序，使其能够适应检查点机制。

检查点同样被用于内存数据库，以提供持久性，而无需在每次系统崩溃后重新处理整个日志。对于内存数据库，一种实用的检查点流程是：创建完整数据库的快照，将其写入两个交替（用于全有或全无原子性）专用非易失性存储区域之一，随后记录一个包含最新快照地址的CHECKPOINT记录。恢复时则需扫描日志至最近的CHECKPOINT记录，收集已提交的全有或全无操作列表，还原该记录描述的快照，然后从CHECKPOINT 记录到日志末尾执行这些已提交操作的重做操作。此场景中的主要挑战在于处理与快照写入并发的更新活动。应对这一挑战可通过在快照期间阻止所有更新，或采用更复杂的前后原子性技术（如本章后续章节所述）来实现。

9.3.6 What if the Cache is not Write-Through? (Advanced Topic)

Between the log and the write-through cache, the logging configurations just described require, for every data update, two synchronous writes to non-volatile storage, with attendant delays waiting for the writes to complete. Since the original reason for introducing a log was to increase performance, these two synchronous write delays usually become the system performance bottleneck. Designers who are interested in maximizing performance would prefer to use a cache that is not write-through, so that writes can be deferred until a convenient time when they can be done in batches. Unfortunately, the application then loses control of the order in which things are actually written to non-volatile storage. Loss of control of order has a significant impact on our all-or-nothing atomicity algorithms, since they require, for correctness, constraints on the order of writes and certainty about which writes have been done.

The first concern is for the log itself because the write-ahead log protocol requires that appending a CHANGE record to the log precede the corresponding install in cell storage. One simple way to enforce the WAL protocol is to make just log writes write-through, but allow cell storage writes to occur whenever the cache manager finds it convenient. However, this relaxation means that if the system crashes there is no assurance that any particular install has actually migrated to non-volatile storage. The recovery procedure, assuming the worst, cannot take advantage of checkpoints and must again perform installs starting from the beginning of the log. To avoid that possibility, the usual design response is to flush the cache as part of logging each checkpoint record. Unfortunately, flushing the cache and logging the checkpoint must be done as a before-or-after action to avoid getting tangled with concurrent updates, which creates another design challenge. This challenge is surmountable, but the complexity is increasing.

Some systems pursue performance even farther. A popular technique is to write the log to a volatile buffer, and *force* that entire buffer to non-volatile storage only when an all-or-nothing action commits. This strategy allows batching several CHANGE records with the next OUTCOME record in a single synchronous write. Although this step would appear to violate the write-ahead log protocol, that protocol can be restored by making the cache used for cell storage a bit more elaborate; its management algorithm must avoid writing back any install for which the corresponding log record is still in the volatile buffer. The trick is to *number* each log record in sequence, and tag each record in the cell storage cache with the sequence number of its log record. Whenever the system forces the log, it tells the cache manager the sequence number of the last log record that it wrote, and the cache manager is careful never to write back any cache record that is tagged with a higher log sequence number.

We have in this section seen some good examples of the *law of diminishing returns* at work: schemes that improve performance sometimes require significantly increased complexity. Before undertaking any such scheme, it is essential to evaluate carefully how much extra performance one stands to gain.

9.3.6 如果缓存不是直写式会怎样？（高级主题）

在日志与直写缓存之间，上述日志配置方案要求每次数据更新时，必须执行两次对非易失性存储的同步写入操作，并伴随等待写入完成的延迟。由于引入日志的初衷是提升性能，这两次同步写入延迟往往成为系统性能瓶颈。追求性能最大化的设计者更倾向于使用非直写式缓存，从而将写入操作推迟至合适时机批量执行。然而，这会导致应用程序丧失对数据实际写入非易失性存储顺序的控制权。写入顺序失控对我们的"全有或全无"原子性算法影响重大——因为这些算法的正确性依赖于对写入顺序的严格约束，以及确认哪些写入已完成的确切信息。

首要关注的是日志本身，因为预写日志协议要求将CHANGE记录追加到日志的操作必须先于对应数据在单元存储中的安装。强制执行WAL协议的一个简单方法是仅使日志写入采用直写模式，而允许单元存储的写入在缓存管理器认为方便时进行。然而，这种宽松处理意味着若系统崩溃，无法保证任何特定安装已实际迁移至非易失性存储。在最坏情况下，恢复过程无法利用检查点，必须从日志起始处重新执行安装操作。为避免这种情况，常规设计对策是在记录每个检查点时刷新缓存作为日志记录的一部分。遗憾的是，为避免与并发更新产生冲突，缓存刷新与检查点记录必须作为原子操作完成，这又带来了新的设计挑战。该挑战虽可克服，但系统复杂度正持续攀升。

有些系统对性能的追求更进一步。一种常见的技术是将日志写入易失性缓冲区，仅当执行全有或全无的提交动作时，才将整个缓冲区*force*到非易失性存储中。这一策略允许将多条CHANGE记录与下一条OUTCOME记录批量处理，通过一次同步写入完成。尽管这一步骤看似违背了预写日志协议，但通过稍微细化用于单元存储的缓存设计即可恢复协议合规性——其管理算法必须避免回写任何日志记录仍驻留于易失性缓冲区的安装操作。关键在于按顺序*number*每条日志记录，并在单元存储缓存中为每条记录标记其对应的日志序列号。每当系统强制刷写日志时，会向缓存管理器通报已写入的最后日志序列号，而缓存管理器则确保绝不回写任何标记着更高日志序列号的缓存记录。

在本节中，我们看到了*law of diminishing returns*发挥作用的一些优秀范例：那些提升性能的方案有时会显著增加复杂性。在实施任何此类方案之前，必须仔细评估能够获得多少额外的性能提升。

9.4 Before-or-After Atomicity I: Concepts

The mechanisms developed in the previous sections of this chapter provide atomicity in the face of failure, so that other atomic actions that take place after the failure and subsequent recovery find that an interrupted atomic action apparently either executed all of its steps or none of them. This and the next section investigate how to also provide atomicity of concurrent actions, known as *before-or-after atomicity*. In this development we will provide *both* all-or-nothing atomicity *and* before-or-after atomicity, so we will now be able to call the resulting atomic actions *transactions*.

Concurrency atomicity requires additional mechanism because when an atomic action installs data in cell storage, that data is immediately visible to all concurrent actions. Even though the version history mechanism can hide pending changes from concurrent atomic actions, they can read other variables that the first atomic action plans to change. Thus, the composite nature of a multiple-step atomic action may still be discovered by a concurrent atomic action that happens to look at the value of a variable in the midst of execution of the first atomic action. Thus, making a composite action atomic with respect to concurrent threads—that is, making it a *before-or-after action*—requires further effort.

Recall that Section 9.1.5 defined the operation of concurrent actions to be correct *if every result is guaranteed to be one that could have been obtained by some purely serial application* of those same actions. So we are looking for techniques that guarantee to produce the same result as if concurrent actions had been applied serially, yet maximize the performance that can be achieved by allowing concurrency.

In this Section 9.4 we explore three successively better before-or-after atomicity schemes, where "better" means that the scheme allows more concurrency. To illustrate the concepts we return to version histories, which allow a straightforward and compelling correctness argument for each scheme. Because version histories are rarely used in practice, in the following Section 9.5 we examine a somewhat different approach, locks, which are widely used because they can provide higher performance, but for which correctness arguments are more difficult.

9.4.1 Achieving Before-or-After Atomicity: Simple Serialization

A version history assigns a unique identifier to each atomic action so that it can link tentative versions of variables to the action's outcome record. Suppose that we require that the unique identifiers be consecutive integers, which we interpret as serial numbers, and we modify the procedure BEGIN_TRANSACTION by adding enforcement of the following *simple serialization* rule: each newly created transaction n must, before reading or writing any data, wait until the preceding transaction $n - 1$ has either committed or aborted. (To ensure that there is always a transaction $n - 1$, assume that the system was initialized by creating a transaction number zero with an OUTCOME record in the committed state.) Figure 9.25 shows this version of BEGIN_TRANSACTION. The scheme forces all transactions to execute in the serial order that threads happen to invoke BEGIN_TRANSACTION. Since that

9.4 前后原子性 I：概念

本章前几节所开发的机制确保了在故障情况下的原子性，使得故障发生及后续恢复后执行的其他原子操作能够观察到，一个被中断的原子操作要么看似执行了其所有步骤，要么看似未执行任何步骤。本节及下一节将探讨如何进一步实现并发操作的原子性，即*before-or-after atomicity*。在这一进展中，我们将提供*both*全有或全无的原子性*and*前后一致的原子性，因此我们现在能够将最终实现的原子操作称为*transactions*。

并发原子性需要额外的机制，因为当一个原子操作将数据存入单元存储时，该数据会立即对所有并发操作可见。尽管版本历史机制能够向并发原子操作隐藏待定更改，但这些操作仍可能读取第一个原子操作计划更改的其他变量。因此，一个多步骤原子操作的组合性质仍可能被恰好在其执行过程中查看变量值的并发原子操作所察觉。因此，要使一个组合操作相对于并发线程具有原子性——即将其转化为*before-or-after action*——还需要进一步的努力。

回想一下，9.1.5节定义了并发操作的正确性为这些相同操作的*if every result is guaranteed to be one that could have been obtained by some purely serial application*。因此，我们寻求的技术既要保证产生与串行应用并发操作相同的结果，又要通过允许并发最大化可实现的性能。

在第9.4节中，我们将探讨三种逐步优化的前后原子性方案，这里的“优化”指的是方案能支持更高的并发性。为了阐明这些概念，我们重新审视版本历史——它为每种方案提供了直观且有力的正确性论证。由于版本历史在实际中较少应用，接下来的9.5节我们将研究一种略有不同的方法：锁机制。锁因其能提供更高性能而被广泛采用，但其正确性论证则更为复杂。

9.4.1 实现前后原子性：简单序列化

版本历史为每个原子操作分配唯一标识符，以便将变量的暂定版本链接到该操作的结果记录。假设我们要求这些唯一标识符必须是连续的整数（我们将其解释为序列号），并通过添加以下*sim-ple serialization*规则来修改BEGIN_TRANSACTION流程：每个新创建的事务n在读取或写入任何数据之前，必须等待前一个事务$n-1$提交或中止。（为确保始终存在事务$n-1$，假定系统初始化时创建了编号为零的事务，其OUTCOME记录处于已提交状态。）图9.25展示了BEGIN_TRANSACTION的这一版本。该方案强制所有事务按照线程调用BEGIN_TRANSACTION的串行顺序执行。由于这一机制

```
1  procedure BEGIN_TRANSACTION ()
2      id ← NEW_OUTCOME_RECORD (PENDING)          // Create, initialize, assign id.
3      previous_id ← id – 1
4      wait until previous_id.outcome_record.state ≠ PENDING
5      return id
```

FIGURE 9.25

BEGIN_TRANSACTION with the simple serialization discipline to achieve before-or-after atomicity. In order that there be an *id* – 1 for every value of *id*, startup of the system must include creating a dummy transaction with *id* = 0 and *id.outcome_record.state* set to COMMITTED. Pseudocode for the procedure NEW_OUTCOME_RECORD appears in Figure 9.30.

order is a possible serial order of the various transactions, by definition simple serialization will produce transactions that are serialized and thus are correct before-or-after actions. Simple serialization trivially provides before-or-after atomicity, and the transaction is still all-or-nothing, so the transaction is now atomic both in the case of failure and in the presence of concurrency.

Simple serialization provides before-or-after atomicity by being too conservative: it prevents all concurrency among transactions, even if they would not interfere with one another. Nevertheless, this approach actually has some practical value—in some applications it may be just the right thing to do, on the basis of simplicity. Concurrent threads can do much of their work in parallel because simple serialization comes into play only during those times that threads are executing transactions, which they generally would be only at the moments they are working with shared variables. If such moments are infrequent or if the actions that need before-or-after atomicity all modify the same small set of shared variables, simple serialization is likely to be just about as effective as any other scheme. In addition, by looking carefully at why it works, we can discover less conservative approaches that allow more concurrency, yet still have compelling arguments that they preserve correctness. Put another way, the remainder of study of before-or-after atomicity techniques is fundamentally nothing but invention and analysis of increasingly effective—and increasingly complex—performance improvement measures.

The version history provides a useful representation for this analysis. Figure 9.26 illustrates in a single figure the version histories of a banking system consisting of four accounts named *A*, *B*, *C*, and *D*, during the execution of six transactions, with serial numbers 1 through 6. The first transaction initializes all the objects to contain the value 0 and the following transactions transfer various amounts back and forth between pairs of accounts.

This figure provides a straightforward interpretation of why simple serialization works correctly. Consider transaction 3, which must read and write objects *B* and *C* in order to transfer funds from one to the other. The way for transaction 3 to produce results as if it ran after transaction 2 is for all of 3's input objects to have values that include all the effects of transaction 2—if transaction 2 commits, then any objects it

```
1  procedure BEGIN_TRANSACTION ()
2      id ← NEW_OUTCOME_RECORD (PENDING)              // Create, initialize, assign id.
3      previous_id ← id - 1                           return id
4      wait until previous_id.outcome_record.state ≠ PENDING
```

图9.25

使用简单的串行化规则来实现BEGIN_TRANSACTION的前后原子性。为了确保每个*id*值都对应一个*id* – 1，系统启动时必须创建一个虚拟事务，其中*id* = 0且*id.outcome_record.state*设置为COMMITTED。图9.30展示了过程NEW_OUTCOME_RECORD的伪代码实现。

顺序是各种事务可能的一种串行序列，根据定义，简单串行化将产生被串行化的事务，从而成为正确的前后操作。简单串行化自然而然地提供了前后原子性，且事务仍保持全有或全无的特性，因此无论在故障情况下还是在并发存在时，事务现在都具有原子性。

简单串行化通过过于保守的方式提供前后原子性：它阻止了事务间的所有并发，即使这些事务实际上不会相互干扰。尽管如此，这种方法确实具有一定的实用价值——在某些应用中，基于简洁性的考量，它可能恰恰是最合适的选择。并发线程可以在大部分时间里并行工作，因为简单串行化仅在线程执行事务时发挥作用，而通常这些时刻仅限于它们操作共享变量的时候。如果这类时刻并不频繁，或者需要前后原子性的操作都修改同一小组共享变量，那么简单串行化的效果很可能与其他任何方案不相上下。此外，通过仔细研究其运作原理，我们可以发现更为宽松的方法，这些方法允许更高的并发度，同时仍能有力论证其正确性得以保持。换言之，对前后原子性技术后续研究的本质，无非是发明并分析那些日益高效——也日益复杂——的性能优化措施。

版本历史为这一分析提供了有用的表现形式。图9.26在一张图中展示了由四个账户（分别命名为*A*、*B*、*C*和*D*）组成的银行系统，在执行六个序列号为1至6的交易过程中的版本历史。首个交易将所有对象初始化为0值，随后的交易则在成对账户之间来回转移不同金额。

该图直观地解释了为何简单的串行化能正确运作。以事务3为例，它必须按顺序读取和写入对象*B*与*C*以完成资金转账。若要使事务3产生如同在事务2之后执行的效果，其所有输入对象的值必须包含事务2的全部影响——即当事务2提交时，它所修改的任何对象

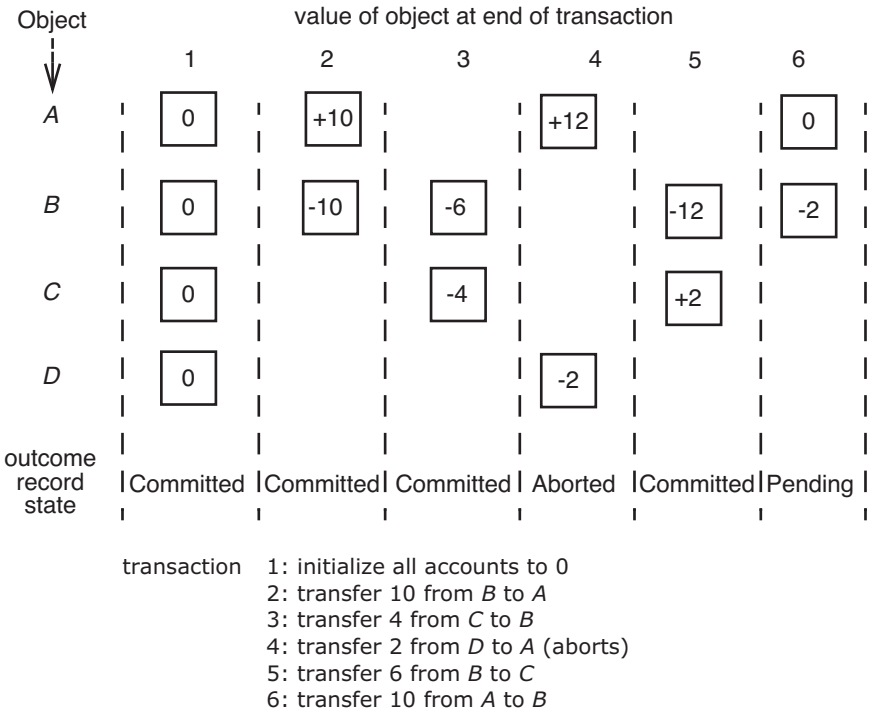

FIGURE 9.26

Version history of a banking system.

changed and that 3 uses should have new values; if transaction 2 aborts, then any objects it tentatively changed and 3 uses should contain the values that they had when transaction 2 started. Since in this example transaction 3 reads *B* and transaction 2 creates a new version of *B*, it is clear that for transaction 3 to produce a correct result it must wait until transaction 2 either commits or aborts. Simple serialization requires that wait, and thus ensures correctness.

Figure 9.26 also provides some clues about how to increase concurrency. Looking at transaction 4 (the example shows that transaction 4 will ultimately abort for some reason, but suppose we are just starting transaction 4 and don't know that yet), it is apparent that simple serialization is too strict. Transaction 4 reads values only from *A* and *D*, yet transaction 3 has no interest in either object. Thus the values of *A* and *D* will be the same whether or not transaction 3 commits, and a discipline that forces 4 to wait for 3's completion delays 4 unnecessarily. On the other hand, transaction 4 does use an object that transaction 2 modifies, so transaction 4 must wait for transaction 2 to complete. Of course, simple serialization guarantees that, since transaction 4 can't begin till transaction 3 completes and transaction 3 couldn't have started until transaction 2 completed.

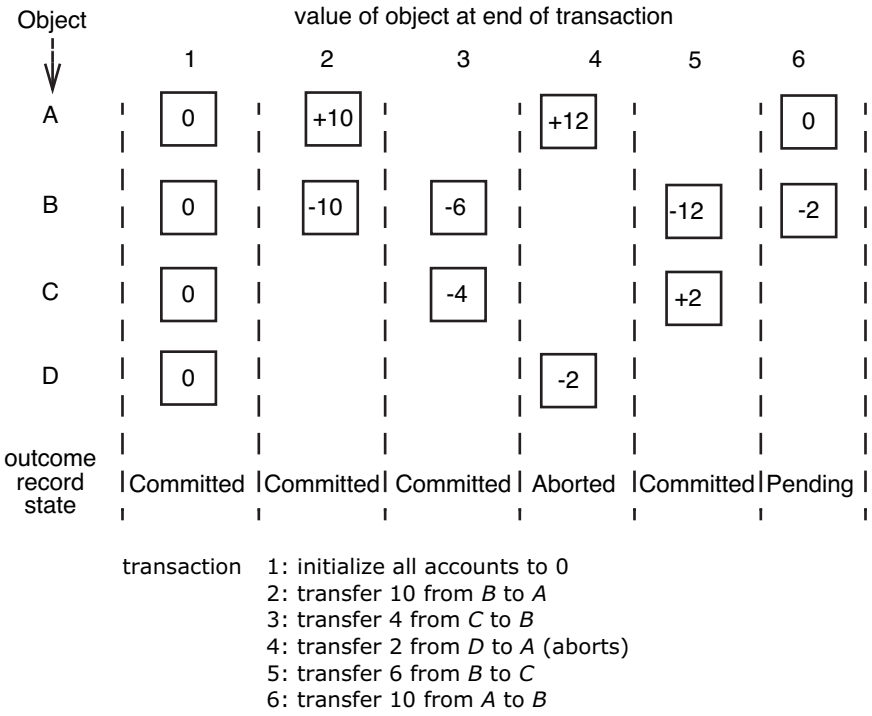

图9.26

银行系统的版本历史。

发生了变化，且3个使用应该具有新的值；如果事务2中止，则它暂态修改的任何对象及3个使用应包含事务2开始时的原有值。在此例中，由于事务3读取了*B*而事务2创建了*B*的新版本，显然事务3要产生正确结果，必须等待事务2提交或中止。简单的序列化要求这一等待，从而确保了正确性。

图9.26也为如何提高并发性提供了一些线索。观察事务4（示例显示事务4最终会因某些原因中止，但假设我们刚开始事务4且尚未知晓该结果），显然简单的串行化过于严格。事务4仅从*A*和*D*读取数值，而事务3对这两个对象均无涉及。因此无论事务3是否提交，*A*和*D*的数值都将保持不变，强制要求事务4等待事务3完成的规则会导致事务4被不必要地延迟。另一方面，事务4确实使用了事务2修改的对象，因此事务4必须等待事务2完成。当然，简单串行化能确保这一点——因为事务4需待事务3完成后才能开始，而事务3又必须等待事务2完成后才能启动。

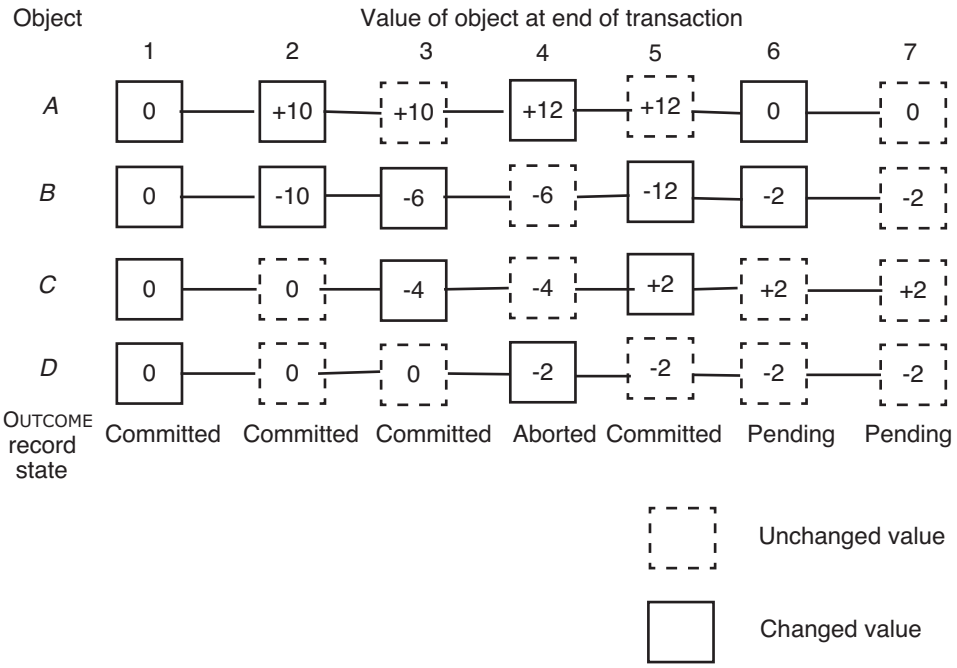

FIGURE 9.27

System state history with unchanged values shown.

These observations suggest that there may be other, more relaxed, disciplines that can still guarantee correct results. They also suggest that any such discipline will probably involve detailed examination of exactly which objects each transaction reads and writes.

Figure 9.26 represents the state history of the entire system in serialization order, but the slightly different representation of Figure 9.27 makes that state history more explicit. In Figure 9.27 it appears that each transaction has perversely created a new version of every object, with unchanged values in dotted boxes for those objects it did not actually change. This representation emphasizes that the vertical slot for, say, transaction 3 is in effect a reservation in the state history for every object in the system; transaction 3 has an opportunity to propose a new value for any object, if it so wishes.

The reason that the system state history is helpful to the discussion is that as long as we eventually end up with a state history that has the values in the boxes as shown, the actual order in real time in which individual object values are placed in those boxes is unimportant. For example, in Figure 9.27, transaction 3 could create its new version of object *C* before transaction 2 creates its new version of *B*. We don't care when things happen, as long as the result is to fill in the history with the same set of values that would result from strictly following this serial ordering. Making the actual time sequence unim-

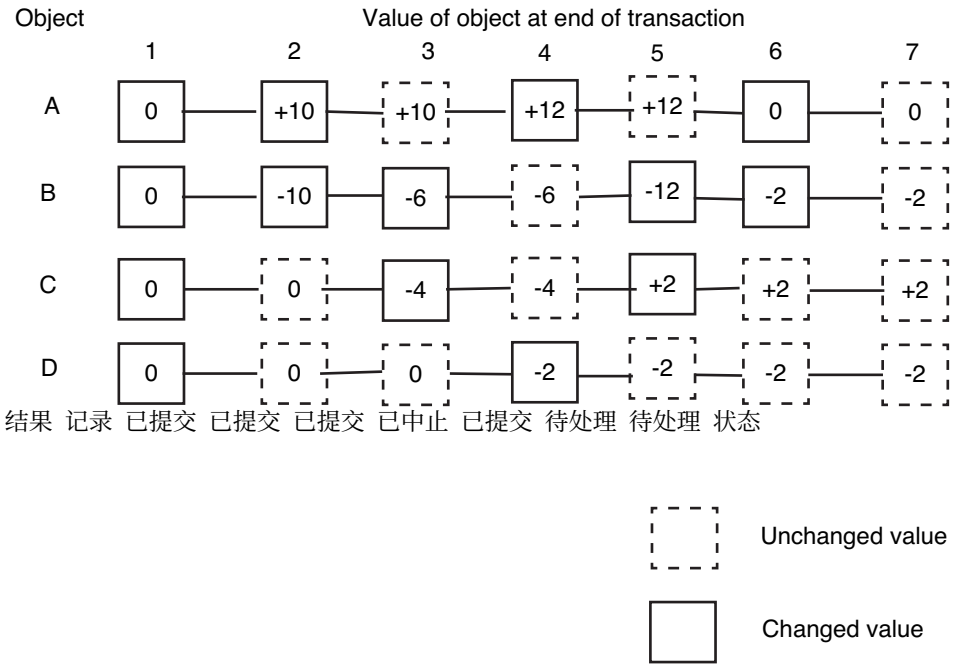

图9.27

系统状态历史，未更改的值已显示。

这些观察表明，可能存在其他更为宽松的规则，仍能确保结果的正确性。它们还暗示，任何此类规则很可能都需要详细检查每个事务具体读取和写入哪些对象。

图9.26以序列化顺序呈现了整个系统的状态历史，而图9.27稍有不同的表现形式使该状态历史更为显化。在图9.27中，看似每笔事务都反常地为每个对象创建了新版本{v*}，其中未实际更改的对象以虚线框标示原值不变。这种表示方式强调，例如事务3的垂直槽位实质上是为系统中每个对象预留的状态历史记录位；事务3有权根据自身需求为任意对象提议新值。

系统状态历史之所以对讨论有帮助，是因为只要最终得到的状态历史中，各框内的值与所示一致，那么各个对象值实际被放入这些框中的实时顺序并不重要。例如，在图9.27中，事务3可以先于事务2创建对象*C*的新版本，而非*B*的新版本。我们并不关心事件发生的具体时机，只要最终填充历史的结果与严格遵循这一串行顺序所得到的值集合相同即可。这使得实际时间序列变得无关紧要——

portant is exactly our goal, since that allows us to put concurrent threads to work on the various transactions. There are, of course, constraints on time ordering, but they become evident by examining the state history.

Figure 9.27 allows us to see just what time constraints must be observed in order for the system state history to record this particular sequence of transactions. In order for a transaction to generate results appropriate for its position in the sequence, it should use as its input values the latest versions of all of its inputs. If Figure 9.27 were available, transaction 4 could scan back along the histories of its inputs *A* and *D*, to the most recent solid boxes (the ones created by transactions 2 and 1, respectively) and correctly conclude that if transactions 2 and 1 have committed then transaction 4 can proceed—even if transaction 3 hasn't gotten around to filling in values for *B* and *C* and hasn't decided whether or not it should commit.

This observation suggests that any transaction has enough information to ensure before-or-after atomicity with respect to other transactions if it can discover the dotted-versus-solid status of those version history boxes to its left. The observation also leads to a specific before-or-after atomicity discipline that will ensure correctness. We call this discipline *mark-point*.

9.4.2 The Mark-Point Discipline

Concurrent threads that invoke READ_CURRENT_VALUE as implemented in Figure 9.15 can not see a pending version of any variable. That observation is useful in designing a before-or-after atomicity discipline because it allows a transaction to reveal all of its results at once simply by changing the value of its OUTCOME record to COMMITTED. But in addition to that we need a way for later transactions that need to read a pending version to wait for it to become committed. The way to do that is to modify READ_CURRENT_VALUE to wait for, rather than skip over, pending versions created by transactions that are earlier in the sequential ordering (that is, they have a smaller *caller_id*), as implemented in lines 4–9 of Figure 9.28. Because, with concurrency, a transaction later in the ordering may create a new version of the same variable before this transaction reads it, READ_CURRENT_VALUE still skips over any versions created by transactions that have a larger *caller_id*. Also, as before, it may be convenient to have a READ_MY_VALUE procedure (not shown) that returns pending values previously written by the running transaction.

Adding the ability to wait for pending versions in READ_CURRENT_VALUE is the first step; to ensure correct before-or-after atomicity we also need to arrange that all variables that a transaction needs as inputs, but that earlier, not-yet-committed transactions plan to modify, have pending versions. To do that we call on the application programmer (for example, the programmer of the TRANSFER transaction) do a bit of extra work: each transaction should create new, pending versions of every variable it intends to modify, and announce when it is finished doing so. Creating a pending version has the effect of marking those variables that are not ready for reading by later transactions, so we will call the point at which a transaction has created them all the *mark point* of the transaction. The

重要之处正是我们的目标，因为这使我们能够让并发线程处理各种事务。当然，时间顺序上存在约束，但通过检查状态历史，这些约束会变得显而易见。

图9.27让我们能够看清系统状态历史记录这一特定交易序列时必须遵守的时间约束。为了使其交易结果符合序列中的位置要求，每笔交易都应使用其所有输入项的最新版本{v*}作为输入值。若图9.27可用，交易4可沿其输入项*A*和*D*的历史记录回溯至最近的实心方框（分别由交易2和1创建），从而正确判定：只要交易2和1已提交，即使交易3尚未完成对*B*和*C*的赋值且未决定是否提交，交易4仍可继续执行。

这一观察表明，任何事务只要能够发现其左侧版本历史框的虚线或实线状态，就拥有足够信息来确保相对于其他事务的先后原子性。该观察还引出了一个具体的先后原子性规则，该规则将确保正确性。我们称这一规则为*mark-point*。

9.4.2 标记点规则

并发线程调用图9.15中实现的READ_CURRENT_VALUE时，无法看到任何变量的待定版本。这一观察在设计前后原子性规则时非常有用，因为它允许事务通过简单地将其OUTCOME记录的值更改为COMMITTED来一次性展示所有结果。但除此之外，我们还需要一种方法让后续需要读取待定版本的事务能够等待其变为已提交状态。实现这一目标的方法是修改READ_CURRENT_VALUE，使其等待而非跳过由顺序排序中较早事务（即具有较小*caller_id*的事务）创建的待定版本，如图9.28第4-9行所示。由于在并发情况下，排序较后的事务可能在该事务读取同一变量之前创建其新版本，READ_CURRENT_VALUE仍会跳过任何由具有较大*caller_id*的事务创建的版本。此外，与之前一样，提供一个READ_MY_VALUE过程（未展示）可能较为便利，该过程可返回当前事务先前写入的待定值。

在READ_CURRENT_VALUE中添加等待待定版本的能力是第一步；为确保正确的前后原子性，我们还需安排事务所需的所有输入变量——那些早先未提交事务计划修改的变量——都拥有待定版本。为此，我们要求应用程序员（例如编写TRANSFER事务的程序员）额外做一些工作：每个事务应为其计划修改的每个变量创建新的待定版本，并在完成时予以声明。创建待定版本的效果是标记这些变量暂不供后续事务读取，因此我们将事务完成创建所有待定版本的时间点称为该事务的*mark point*。

transaction announces that it has passed its mark point by calling a procedure named MARK_POINT_ANNOUNCE, which simply sets a flag in the outcome record for that transaction.

The mark-point discipline then is that no transaction can begin reading its inputs until the preceding transaction has reached its mark point or is no longer pending. This discipline requires that each transaction identify which data it will update. If the transaction has to modify some data objects before it can discover the identity of others that require update, it could either delay setting its mark point until it does know all of the objects it will write (which would, of course, also delay all succeeding transactions) or use the more complex discipline described in the next section.

For example, in Figure 9.27, the boxes under newly arrived transaction 7 are all dotted; transaction 7 should begin by marking the ones that it plans to make solid. For convenience in marking, we split the WRITE_NEW_VALUE procedure of Figure 9.15 into two parts, named NEW_VERSION and WRITE_VALUE, as in Figure 9.29. Marking then consists simply of a series of calls to NEW_VERSION. When finished marking, the transaction calls MARK_POINT_ANNOUNCE. It may then go about its business, reading and writing values as appropriate to its purpose.

Finally, we enforce the mark point discipline by putting a test and, depending on its outcome, a wait in BEGIN_TRANSACTION, as in Figure 9.30, so that no transaction may begin execution until the preceding transaction either reports that it has reached its mark point or is no longer PENDING. Figure 9.30 also illustrates an implementation of MARK_POINT_ANNOUNCE. No changes are needed in procedures ABORT and COMMIT as shown in Figure 9.13, so they are not repeated here.

Because no transaction can start until the previous transaction reaches its mark point, all transactions earlier in the serial ordering must also have passed their mark points, so every transaction earlier in the serial ordering has already created all of the versions that it ever will. Since READ_CURRENT_VALUE now waits for earlier, pending values to become

```
1  procedure READ_CURRENT_VALUE (data_id, this_transaction_id)
2      starting at end of data_id repeat until beginning
3          v ← previous version of data_id
4          last_modifier ← v.action_id
5          if last_modifier ≥ this_transaction_id then skip v        // Keep searching
6          wait until (last_modifier.outcome_record.state ≠ PENDING)
7          if (last_modifier.outcome_record.state = COMMITTED)
8              then return v.state
9              else skip v                                           // Resume search
10     signal ("Tried to read an uninitialized variable")
```

FIGURE 9.28

READ_CURRENT_VALUE for the mark-point discipline. This form of the procedure skips all versions created by transactions later than the calling transaction, and it waits for a pending version created by an earlier transaction until that earlier transaction commits or aborts.

事务通过调用名为MARK_POINT_ANNOUNCE的过程来宣布其已通过标记点，该过程仅在该事务的结果记录中设置一个标志.

标记点规则即是：在前一事务达到其标记点或不再处于挂起状态之前，任何事务都不能开始读取其输入数据。这一规则要求每个事务必须明确标识它将更新的数据。如果事务在发现其他需要更新的数据对象身份之前必须先修改某些数据对象，它可以选择延迟设置其标记点，直到确实知晓所有待写入对象为止（这当然也会延迟所有后续事务），或者采用下一节所述的更为复杂的规则。

例如，在图9.27中，新到达事务7下方的方框均为虚线标注；事务7应首先标记其计划转为实线的部分。为便于标记操作，我们将图9.15中的WRITE_NEW_VALUE流程拆分为两部分，分别命名为NEW_VERSION和WRITE_VALUE，如图9.29所示。标记过程简化为一系列对NEW_VERSION的调用。完成标记后，事务调用MARK_POINT_ANNOUNCE。随后它便可执行其业务逻辑，根据目标需求进行数值的读写操作。

最后，我们通过在图9.30所示的BEGIN_TRANSACTION处设置一个测试并根据其结果决定是否等待，来强制执行标记点规则。这样，任何事务在前一事务报告其已达到标记点或不再处于PENDING状态之前，都不能开始执行。图9.30还展示了MARK_POINT_ANNOUNCE的一种实现方式。如图9.13所示，过程ABORT和COMMIT无需任何修改，因此此处不再重复。

由于任何事务都必须在前一事务达到其标记点后才能开始，因此序列顺序中较早的所有事务也必然已经通过了它们的标记点，这意味着序列顺序中较早的每个事务都已经创建了它将生成的所有版本。既然READ_CURRENT_VALUE现在等待较早的待定值变为

```
1  procedure READ_CURRENT_VALUE (data_id, this_transaction_id)
2      starting at end of data_id repeat until beginning
3          v ← previous version of data_id
4          last_modifier ← v.action_id
5          if last_modifier ≥ this_transaction_id then skip v        // Keep searching
6          wait until (last_modifier.outcome_record.state ≠ PENDING)
7          if (last_modifier.outcome_record.state = COMMITTED)
8              then return v.state
9              else skip v                                           // Resume search
10     signal ("Tried to read an uninitialized variable")
```

图9.28

READ_CURRENT_VALUE for the mark-point discipline. This form of the procedure skips all versions created by transactions later than the calling transaction, and it waits for a pending version created by an earlier transaction until that earlier transaction commits or aborts.

```
1  procedure NEW_VERSION (reference data_id, this_transaction_id)
2      if this_transaction_id.outcome_record.mark_state = MARKED then
3          signal ("Tried to create new version after announcing mark point!")
4      append new version v to data_id
5      v.value ← NULL
6      v.action_id ← transaction_id

7  procedure WRITE_VALUE (reference data_id, new_value, this_transaction_id)
8      starting at end of data_id repeat until beginning
9          v ← previous version of data_id
10         if v.action_id = this_transaction_id
11             v.value ← new_value
12             return
13     signal ("Tried to write without creating new version!"))
```

FIGURE 9.29

Mark-point discipline versions of NEW_VERSION and WRITE_VALUE.

```
1  procedure BEGIN_TRANSACTION ()
2      id ← NEW_OUTCOME_RECORD (PENDING)
3      previous_id ← id - 1
4      wait until (previous_id.outcome_record.mark_state = MARKED)
5          or (previous_id.outcome_record.state ≠ PENDING)
6      return id

7  procedure NEW_OUTCOME_RECORD (starting_state)
8      ACQUIRE (outcome_record_lock)        // Make this a before-or-after action.
9      id ← TICKET (outcome_record_sequencer)
10     allocate id.outcome_record
11     id.outcome_record.state ← starting_state
12     id.outcome_record.mark_state ← NULL
13     RELEASE (outcome_record_lock)
14     return id

15 procedure MARK_POINT_ANNOUNCE (reference this_transaction_id)
16     this_transaction_id.outcome_record.mark_state ← MARKED
```

FIGURE 9.30

The procedures BEGIN_TRANSACTION, NEW_OUTCOME_RECORD, and MARK_POINT_ANNOUNCE for the mark-point discipline. BEGIN_TRANSACTION presumes that there is always a preceding transaction. so the system should be initialized by calling NEW_OUTCOME_RECORD to create an empty initial transaction in the *starting_state* COMMITTED and immediately calling MARK_POINT_ANNOUNCE for the empty transaction.

```
1  procedure NEW_VERSION (reference data_id, this_transaction_id)
2      if this_transaction_id.outcome_record.mark_state = MARKED then
3          signal ("Tried to create new version after announcing mark point!")
4      append new version v to data_id
5      v.value ← NULL
6      v.action_id ← transaction_id

7  procedure WRITE_VALUE (reference data_id, new_value, this_transaction_id)
8      starting at end of data_id repeat until beginning
9          v ← previous version of data_id
10         if v.action_id = this_transaction_id
11             v.value ← new_value
12             return
13     signal ("Tried to write without creating new version!"))
```

图9.29

NEW_VERSION 和 WRITE_VALUE 的标记点规范版本。

```
1  procedure BEGIN_TRANSACTION ()
2      id ← NEW_OUTCOME_RECORD (PENDING)
3      previous_id ← id - 1
4      wait until (previous_id.outcome_record.mark_state = MARKED)
5          or (previous_id.outcome_record.state ≠ PENDING)
6      return id

7  procedure NEW_OUTCOME_RECORD (starting_state)
8      ACQUIRE (outcome_record_lock)        // Make this a before-or-after action.
9      id ← TICKET (outcome_record_sequencer)
10     allocate id.outcome_record
11     id.outcome_record.state ← starting_state
12     id.outcome_record.mark_state ← NULL
13     RELEASE (outcome_record_lock)
14     return id

15 procedure MARK_POINT_ANNOUNCE (reference this_transaction_id)
16     this_transaction_id.outcome_record.mark_state ← MARKED
```

图9.30

The procedures BEGIN_TRANSACTION, NEW_OUTCOME_RECORD, and MARK_POINT_ANNOUNCE for the mark-point discipline. BEGIN_TRANSACTION presumes that there is always a preceding transaction. so the system should be initialized by calling NEW_OUTCOME_RECORD to create an empty initial transaction in the *starting_state* COMMITTED and immediately calling MARK_POINT_ANNOUNCE for the empty transaction.

committed or aborted, it will always return to its client a value that represents the final outcome of all preceding transactions. All input values to a transaction thus contain the committed result of all transactions that appear earlier in the serial ordering, just as if it had followed the simple serialization discipline. The result is thus guaranteed to be exactly the same as one produced by a serial ordering, no matter in what real time order the various transactions actually write data values into their version slots. The particular serial ordering that results from this discipline is, as in the case of the simple serialization discipline, the ordering in which the transactions were assigned serial numbers by NEW_OUTCOME_RECORD.

There is one potential interaction between all-or-nothing atomicity and before-or-after atomicity. If pending versions survive system crashes, at restart the system must track down all PENDING transaction records and mark them ABORTED to ensure that future invokers of READ_CURRENT_VALUE do not wait for the completion of transactions that have forever disappeared.

The mark-point discipline provides before-or-after atomicity by bootstrapping from a more primitive before-or-after atomicity mechanism. As usual in bootstrapping, the idea is to reduce some general problem—here, that problem is to provide before-or-after atomicitiy for arbitrary application programs—to a special case that is amenable to a special-case solution—here, the special case is construction and initialization of a new outcome record. The procedure NEW_OUTCOME_RECORD in Figure 9.30 must itself be a before-or-after action because it may be invoked concurrently by several different threads and it must be careful to give out different serial numbers to each of them. It must also create completely initialized outcome records, with *value* and *mark_state* set to PENDING and NULL, respectively, because a concurrent thread may immediately need to look at one of those fields. To achieve before-or-after atomicity, NEW_OUTCOME_RECORD bootstraps from the TICKET procedure of Section 5.6.3 to obtain the next sequential serial number, and it uses ACQUIRE and RELEASE to make its initialization steps a before-or-after action. Those procedures in turn bootstrap from still lower-level before-or-after atomicity mechanisms, so we have three layers of bootstrapping.

We can now reprogram the funds TRANSFER procedure of Figure 9.15 to be atomic under both failure and concurrent activity, as in Figure 9.31. The major change from the earlier version is addition of lines 4 through 6, in which TRANSFER calls NEW_VERSION to mark the two variables that it intends to modify and then calls MARK_POINT_ANNOUNCE. The interesting observation about this program is that most of the work of making actions before-or-after is actually carried out in the called procedures. The only effort or thought required of the application programmer is to identify and mark, by creating new versions, the variables that the transaction will modify.

The delays (which under the simple serialization discipline would all be concentrated in BEGIN_TRANSACTION) are distributed under the mark-point discipline. Some delays may still occur in BEGIN_TRANSACTION, waiting for the preceding transaction to reach its mark point. But if marking is done before any other calculations, transactions are likely to reach their mark points promptly, and thus this delay should be not as great as waiting for them to commit or abort. Delays can also occur at any invocation of

无论事务是提交还是中止，它总会向客户端返回一个代表所有先前事务最终结果的值。因此，事务的所有输入值都包含了序列顺序中早于它的所有事务已提交的结果，就如同遵循了简单的串行化规则一样。这样一来，无论各事务实际以何种实时顺序将数据值写入其版本槽，结果都保证与串行顺序产生的结果完全相同。正如简单串行化规则的情况一样，由此规则产生的特定串行顺序，就是事务被NEW_OUTCOME_RECORD分配序列号的顺序。

全有或全无原子性与前后原子性之间存在一种潜在的交互作用。如果待定版本在系统崩溃后得以保留，那么在重启时，系统必须追踪所有PENDING事务记录并将其标记为ABORTED，以确保READ_CURRENT_VALUE的未来调用者不会等待那些已永久消失的事务完成。

标记点规则通过从一个更原始的前后原子性机制中引导，提供了前后原子性。如同引导过程中的常见做法，其核心思想是将一个普遍问题——此处即为为任意应用程序提供前后原子性——简化为一个适合特殊解决方案的特殊案例。这里的特殊案例是新建并初始化一个结果记录。图9.30中的过程NEW_OUTCOME_RECORD本身必须是一个前后原子操作，因为它可能被多个不同线程并发调用，且必须确保为每个线程分配不同的序列号。同时，它还需创建完全初始化的结果记录，其中*value*和*mark_state*分别设置为PENDING 和NULL，因为并发线程可能立即需要访问这些字段之一。为实现前后原子性，NEW_OUTCOME_RECORD从第5.6.3节的TICKET过程引导获取下一个顺序序列号，并利用ACQUIRE 和RELEASE使其初始化步骤成为前后原子操作。这些过程又进一步从更低层级的前后原子性机制中引导，因此我们共有三层引导机制。

我们现在可以重新编写图9.15中的资金TRANSFER处理流程，使其在图9.31所示情况下同时具备故障原子性和并发原子性。相较于早期版本，主要改动在于新增了第4至6行代码——TRANSFER通过调用NEW_VERSION标记待修改变量，随后调用MARK_POINT_ANNOUNCE。该程序值得注意的特点是：实现动作"前或后"语义的大部分工作实际上由被调用过程完成。应用程序员仅需通过创建新版本标识并标记事务将修改的变量，无需额外投入精力或思考。

在简单串行化规则下会全部集中在BEGIN_TRANSACTION的延迟，在标记点规则下被分散开来。部分延迟仍可能出现在BEGIN_TRANSACTION，等待前一个事务到达其标记点。但如果标记操作先于其他任何计算完成，事务就能迅速抵达各自的标记点，因此这种延迟应该不会像等待事务提交或中止那样严重。延迟也可能发生在任何调用

```
1  procedure TRANSFER (reference debit_account, reference credit_account,
2                                                                amount)
3      my_id ← BEGIN_TRANSACTION ()
4      NEW_VERSION (debit_account, my_id)
5      NEW_VERSION (credit_account, my_id)
6      MARK_POINT_ANNOUNCE (my_id);
7      xvalue ← READ_CURRENT_VALUE (debit_account, my_id)
8      xvalue ← xvalue - amount
9      WRITE_VALUE (debit_account, xvalue, my_id)
10     yvalue ← READ_CURRENT_VALUE (credit_account, my_id)
11     yvalue ← yvalue + amount
12     WRITE_VALUE (credit_account, yvalue, my_id)
13     if xvalue > 0 then
14         COMMIT (my_id)
15     else
16         ABORT (my_id)
17         signal("Negative transfers are not allowed.")
```

FIGURE 9.31

An implementation of the funds transfer procedure that uses the mark point discipline to ensure that it is atomic both with respect to failure and with respect to concurrent activity.

READ_CURRENT_VALUE, but only if there is really something that the transaction must wait for, such as committing a pending version of a necessary input variable. Thus the overall delay for any given transaction should never be more than that imposed by the simple serialization discipline, and one might anticipate that it will often be less.

A useful property of the mark-point discipline is that it never creates deadlocks. Whenever a wait occurs it is a wait for some transaction *earlier* in the serialization. That transaction may in turn be waiting for a still earlier transaction, but since no one ever waits for a transaction later in the ordering, progress is guaranteed. The reason is that at all times there must be some earliest pending transaction. The ordering property guarantees that this earliest pending transaction will encounter no waits for other transactions to complete, so it, at least, can make progress. When it completes, some other transaction in the ordering becomes earliest, and it now can make progress. Eventually, by this argument, every transaction will be able to make progress. This kind of reasoning about progress is a helpful element of a before-or-after atomicity discipline. In Section 9.5 of this chapter we will encounter before-or-after atomicity disciplines that are correct in the sense that they guarantee the same result as a serial ordering, but they do not guarantee progress. Such disciplines require additional mechanisms to ensure that threads do not end up deadlocked, waiting for one another forever.

Two other minor points are worth noting. First, if transactions wait to announce their mark point until they are ready to commit or abort, the mark-point discipline reduces to the simple serialization discipline. That observation confirms that one disci-

```
1  procedure TRANSFER (reference debit_account, reference credit_account,
2                                                                amount)
3      my_id ← BEGIN_TRANSACTION ()
4      NEW_VERSION (debit_account, my_id)
5      NEW_VERSION (credit_account, my_id)
6      MARK_POINT_ANNOUNCE (my_id);
7      xvalue ← READ_CURRENT_VALUE (debit_account, my_id)
8      xvalue ← xvalue - amount
9      WRITE_VALUE (debit_account, xvalue, my_id)
10     yvalue ← READ_CURRENT_VALUE (credit_account, my_id)
11     yvalue ← yvalue + amount
12     WRITE_VALUE (credit_account, yvalue, my_id)
13     if xvalue > 0 then
14         COMMIT (my_id)
15     else
16         ABORT (my_id)
17         signal("Negative transfers are not allowed.")
```

图9.31

一种资金转账程序的实现，采用标记点规则确保其在故障和并发活动方面均具有原子性。

READ_CURRENT_VALUE，但仅当确实存在事务必须等待的内容时才会执行，例如提交必要输入变量的待定版本。因此，任何给定事务的总体延迟不应超过简单串行化规则所施加的限制，并且可以预期，实际延迟往往会更短。

标记点机制的一个有用特性是它永远不会产生死锁。每当出现等待时，等待的对象总是序列化中的某个事务*earlier*。该事务可能又在等待更早的事务，但由于没有人会等待排序中更靠后的事务，因此系统总能保证进展。原因在于，任何时候都必然存在某个最早待处理的事务。排序特性确保这个最早待处理事务不会因等待其他事务完成而受阻，因此它至少能取得进展。当它完成时，排序中的另一个事务就成为最早事务，此时该事务便能推进。通过这种论证，最终每个事务都能取得进展。这种关于系统进展的推理方式，是"前后"原子性机制的有益组成部分。在本章9.5节，我们将探讨一些"前后"原子性机制——它们能保证产生与串行排序相同的结果（在这个意义上是正确的），但无法保证系统进展。这类机制需要额外的方法来确保线程不会陷入永久相互等待的死锁状态。

还有两点次要内容值得注意。首先，如果事务等到准备提交或中止时才宣布其标记点{v*}，那么标记点规则就简化为简单的串行化规则。这一观察证实了某一规

pline is a relaxed version of the other. Second, there are at least two opportunities in the mark-point discipline to discover and report protocol errors to clients. A transaction should never call NEW_VERSION after announcing its mark point. Similarly, WRITE_VALUE can report an error if the client tries to write a value for which a new version was never created. Both of these error-reporting opportunities are implemented in the pseudocode of Figure 9.29.

9.4.3 Optimistic Atomicity: Read-Capture (Advanced Topic)

Both the simple serialization and mark-point disciplines are concurrency control methods that may be described as *pessimistic*. That means that they presume that interference between concurrent transactions is likely and they actively prevent any possibility of interference by imposing waits at any point where interference might occur. In doing so, they also may prevent some concurrency that would have been harmless to correctness. An alternative scheme, called *optimistic* concurrency control, is to presume that interference between concurrent transactions is unlikely, and allow them to proceed without waiting. Then, watch for actual interference, and if it happens take some recovery action, for example aborting an interfering transaction and makaing it restart. (There is a popular tongue-in-cheek characterization of the difference: pessimistic = "ask first", optimistic = "apologize later".) The goal of optimistic concurrency control is to increase concurrency in situations where actual interference is rare.

The system state history of Figure 9.27 suggests an opportunity to be optimistic. We could allow transactions to write values into the system state history in any order and at any time, but with the risk that some attempts to write may be met with the response "Sorry, that write would interfere with another transaction. You must abort, abandon this serialization position in the system state history, obtain a later serialization, and rerun your transaction from the beginning."

A specific example of this approach is the *read-capture* discipline. Under the read-capture discipline, there is an option, but not a requirement, of advance marking. Eliminating the requirement of advance marking has the advantage that a transaction does not need to predict the identity of every object it will update—it can discover the identity of those objects as it works. Instead of advance marking, whenever a transaction calls READ_CURRENT_VALUE, that procedure makes a mark at this thread's position in the version history of the object it read. This mark tells potential version-inserters earlier in the serial ordering but arriving later in real time that they are no longer allowed to insert—they must abort and try again, using a later serial position in the version history. Had the prospective version inserter gotten there sooner, before the reader had left its mark, the new version would have been acceptable, and the reader would have instead waited for the version inserter to commit, and taken that new value instead of the earlier one. Read-capture gives the reader the power of extending validity of a version through intervening transactions, up to the reader's own serialization position. This view of the situation is illustrated in Figure 9.32, which has the same version history as did Figure 9.27.

pline是另一种的宽松版本。其次，在标记点（mark-point）规范中，至少存在两次机会来发现并向客户端报告协议错误。事务在宣布其标记点后绝不应调用NEW_VERSION。同样地，如果客户端尝试写入一个从未创建新版本的值，WRITE_VALUE 可以报告错误。这两种错误报告机会均在图9.29的伪代码中实现。

9.4.3 乐观原子性：读取捕获（高级主题）

简单的串行化和标记点规则都是并发控制方法，可被描述为*pessimistic*。这意味着它们假定并发事务间很可能发生干扰，并通过在可能发生干扰的任何点强制等待，主动防止任何干扰的可能性。这样做时，它们也可能阻止了一些对正确性无害的并发操作。另一种方案称为*optimistic*并发控制，它假定并发事务间的干扰不太可能发生，允许它们无需等待地继续执行。然后，监视实际的干扰情况，如果发生干扰，则采取一些恢复措施，例如中止干扰事务并使其重新启动。（有一种流行的戏谑说法形容两者区别：悲观的=“先问再做”，乐观的=“事后道歉”。）乐观并发控制的目标是在实际干扰罕见的情况下提高并发性。

图9.27所示的系统状态历史表明存在乐观处理的可能。我们可以允许事务以任意顺序、任意时间将数值写入系统状态历史，但需承担某些写入尝试可能遭遇回应的风险："抱歉，该写入会干扰其他事务。您必须中止操作，放弃当前在系统状态历史中的序列化位置，获取更新的序列化点，并从头重新运行您的事务。"

这种方法的一个具体例子是*read-capture*规则。在读取捕获规则下，存在提前标记的选项，但并非强制要求。取消提前标记的要求有一个优势，即事务无需预测它将更新的每个对象的身份——它可以在工作时发现这些对象的身份。替代提前标记的是，每当事务调用READ_CURRENT_VALUE, 时，该过程会在该线程读取对象的版本历史中的当前位置做一个标记。这个标记告诉那些在串行顺序中较早但在实际时间中较晚到达的潜在版本插入者，他们不再被允许插入——必须中止并重试，使用版本历史中更靠后的串行位置。如果预期的版本插入者能更早到达，在读取者留下标记之前，新版本本是可以接受的，而读取者则会等待版本插入者提交，并采用那个新值而非较早的值。读取捕获赋予读取者通过介入事务延长版本有效性的能力，直至读取者自身的串行化位置。这一情境的视图如图9.32所示，其版本历史与图9.27相同。

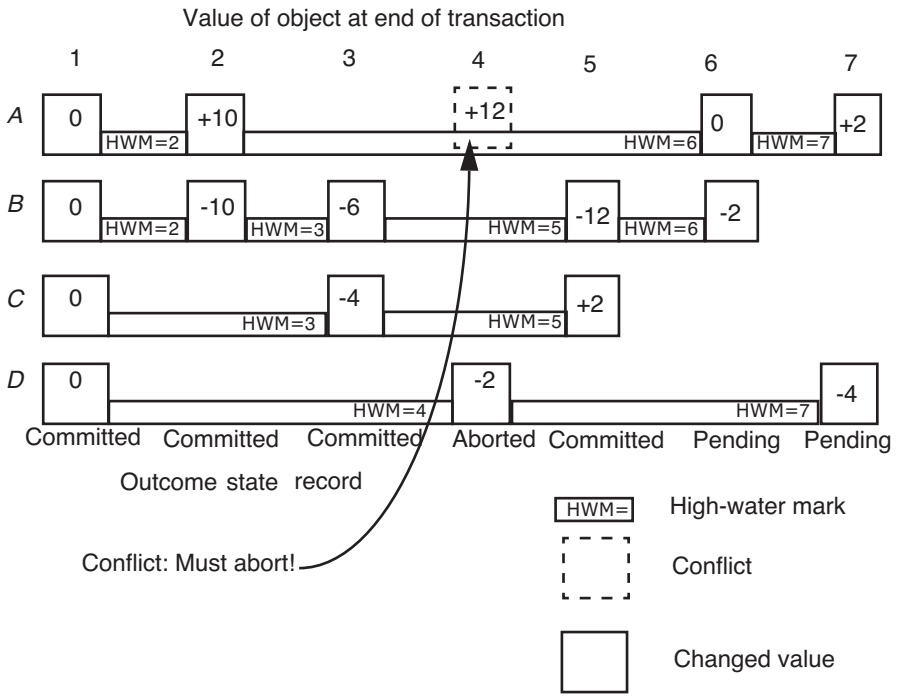

FIGURE 9.32

Version history with high-water marks and the read-capture discipline. First, transaction 6, which is running concurrently with transaction 4, reads variable *A*, thus extending the high-water mark of *A* to 6. Then, transaction 4 (which intends to transfer 2 from *D* to *A*) encounters a conflict when it tries to create a new version of *A* and discovers that the high-water mark of *A* has already been set by transaction 6, so 4 aborts and returns as transaction 7. Transaction 7 retries transaction 4, extending the high-water marks of *A* and *D* to 7.

The key property of read-capture is illustrated by an example in Figure 9.32. Transaction 4 was late in creating a new version of object *A*; by the time it tried to do the insertion, transaction 6 had already read the old value (+10) and thereby extended the validity of that old value to the beginning of transaction 6. Therefore, transaction 4 had to be aborted; it has been reincarnated to try again as transaction 7. In its new position as transaction 7, its first act is to read object *D*, extending the validity of its most recent committed value (zero) to the beginning of transaction 7. When it tries to read object *A*, it discovers that the most recent version is still uncommitted, so it must wait for transaction 6 to either commit or abort. Note that if transaction 6 should now decide to create a new version of object *C*, it can do so without any problem, but if it should try to create a new version of object *D*, it would run into a conflict with the old, now extended version of *D*, and it would have to abort.

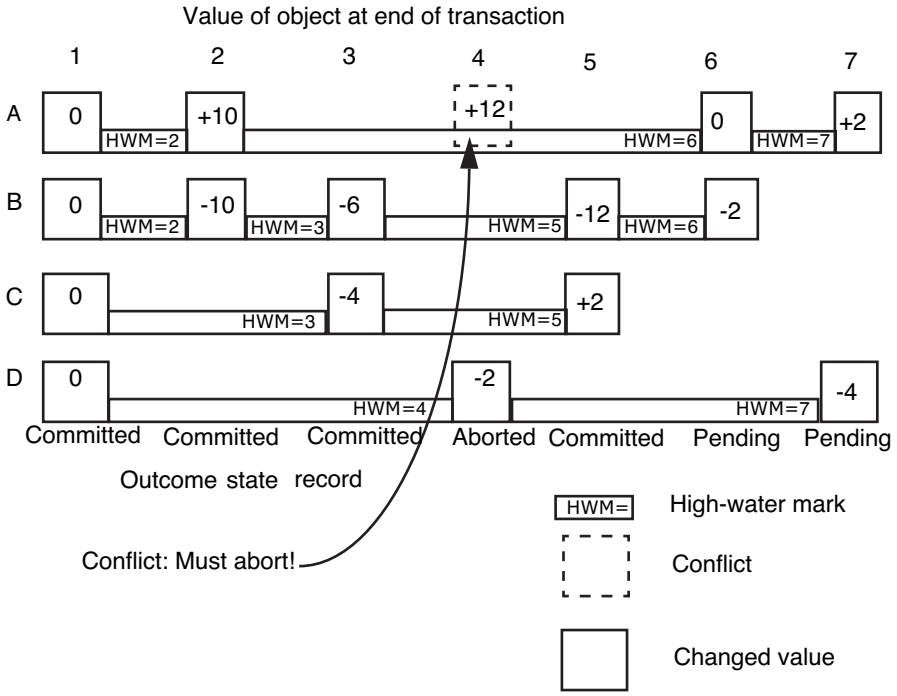

图9.32

带有高水位标记和读取捕获机制的版本历史。首先，与事务4并发运行的事务6读取了变量A，从而将A的高水位标记提升至6。接着，事务4（意图从*D*向*A*转账2）在尝试创建A的新版本时遭遇冲突，发现*A*的高水位标记已被事务6设定，因此事务4中止并以事务7的身份重试。事务7重新执行事务4的操作，将*A*和*D*的高水位标记更新至7。

读取捕获的关键特性通过图9.32中的示例得以阐明。事务4在创建对象*A*的新版本时延迟了；当它尝试执行插入操作时，事务6已经读取了旧值（+10），从而将该旧值的有效期延长至事务6的开始时刻。因此，事务4不得不被中止；它已转生为事务7重新尝试。作为事务7的新身份，其首要操作是读取对象*D*，这将最近提交的值（零）的有效期延伸至事务7的起始点。当它尝试读取对象*A*时，发现最新版本仍未提交，故必须等待事务6提交或中止。值得注意的是，若事务6此时决定创建对象*C*的新版本，可以无障碍执行；但若试图创建对象*D*的新版本，则会与现已延期的旧版本*D*产生冲突，导致事务6必须中止。

```
1   procedure READ_CURRENT_VALUE (reference data_id, value, caller_id)
2       starting at end of data_id repeat until beginning
3           v ← previous version of data_id
4           if v.action_id ≥ caller_id then skip v
5           examine v.action_id.outcome_record
6               if PENDING then
7                   WAIT for v.action_id to COMMIT or ABORT
8                   if COMMITTED then
9                       v.high_water_mark ← max(v.high_water_mark, caller_id)
10                      return v.value
11                  else skip v                          // Continue backward search
12      signal ("Tried to read an uninitialized variable!")

13  procedure NEW_VERSION (reference data_id, caller_id)
14      if (caller_id < data_id.high_water_mark)        // Conflict with later reader.
15      or (caller_id < (LATEST_VERSION[data_id].action_id))     // Blind write conflict.
16      then ABORT this transaction and terminate this thread
17      add new version v at end of data_id
18      v.value ← 0
19      v.action_id ← caller_id

20  procedure WRITE_VALUE (reference data_id, new_value, caller_id)
21      locate version v of data_id.history such that v.action_id = caller_id
22              (if not found, signal ("Tried to write without creating new version!"))
23      v.value ← new_value
```

FIGURE 9.33

Read-capture forms of READ_CURRENT_VALUE, NEW_VERSION, and WRITE_VALUE.

Read-capture is relatively easy to implement in a version history system. We start, as shown in Figure 9.33, by adding a new step (at line 9) to READ_CURRENT_VALUE. This new step records with each data object a *high-water mark*—the serial number of the highest-numbered transaction that has ever read a value from this object's version history. The high-water mark serves as a warning to other transactions that have earlier serial numbers but are late in creating new versions. The warning is that someone later in the serial ordering has already read a version of this object from earlier in the ordering, so it is too late to create a new version now. We guarantee that the warning is heeded by adding a step to NEW_VERSION (at line 14), which checks the high-water mark for the object to be written, to see if any transaction with a higher serial number has already read the current version of the object. If not, we can create a new version without concern. But if the transaction serial number in the high-water mark is greater than this transaction's own serial number, this transaction must abort, obtain a new, higher serial number, and start over again.

```
1   procedure READ_CURRENT_VALUE (reference data_id, value, caller_id)
2       starting at end of data_id repeat until beginning
3           v ← previous version of data_id
4           if v.action_id ≥ caller_id then skip v
5           examine v.action_id.outcome_record
6               if PENDING then
7                   WAIT for v.action_id to COMMIT or ABORT
8                   if COMMITTED then
9                       v.high_water_mark ← max(v.high_water_mark, caller_id)
10                      return v.value
11                  else skip v                          // Continue backward search
12      signal ("Tried to read an uninitialized variable!")

13  procedure NEW_VERSION (reference data_id, caller_id)
14      if (caller_id < data_id.high_water_mark)        // Conflict with later reader.
15      or (caller_id < (LATEST_VERSION[data_id].action_id))     // Blind write conflict.
16      then ABORT this transaction and terminate this thread
17      add new version v at end of data_id
18      v.value ← 0
19      v.action_id ← caller_id

20  procedure WRITE_VALUE (reference data_id, new_value, caller_id)
21      locate version v of data_id.history such that v.action_id = caller_id
22              (if not found, signal ("Tried to write without creating new version!"))
23      v.value ← new_value
```

图9.33

读取捕获形式的READ_CURRENT_VALUE、NEW_VERSION以及WRITE_VALUE。

在版本历史系统中实现读捕获相对容易。如图9.33所示，我们从在READ_CURRENT_VALUE处新增一个步骤（第9行）开始。这个新步骤会为每个数据对象记录一个*high-water mark*——即曾从该对象版本历史中读取过值的最高编号事务的序列号。这个高水位标记相当于对具有更早序列号但延迟创建新版本的其他事务发出警告：序列顺序中更靠后的事务已经读取了该对象在顺序中更早的版本，因此现在创建新版本为时已晚。为确保该警告被遵守，我们在NEW_VERSION处（第14行）增加了一个步骤，检查待写入对象的高水位标记，确认是否有更高序列号的事务已读取该对象的当前版本。如果没有，则可以无忧创建新版本。但如果高水位标记中的事务序列号大于当前事务自身的序列号，则该事务必须中止，获取一个新的更高序列号，并重新开始。

We have removed all constraints on the real-time sequence of the constituent steps of the concurrent transaction, so there is a possibility that a high-numbered transaction will create a new version of some object, and then later a low-numbered transaction will try to create a new version of the same object. Since our NEW_VERSION procedure simply tacks new versions on the end of the object history, we could end up with a history in the wrong order. The simplest way to avoid that mistake is to put an additional test in NEW_VERSION (at line 15), to ensure that every new version has a client serial number that is larger than the serial number of the next previous version. If not, NEW_VERSION aborts the transaction, just as if a read-capture conflict had occurred. (This test aborts only those transactions that perform conflicting *blind writes*, which are uncommon. If either of the conflicting transactions reads the value before writing it, the setting and testing of *high_water_mark* will catch and prevent the conflict.)

The first question one must raise about this kind of algorithm is whether or not it actually works: is the result always the same as some serial ordering of the concurrent transactions? Because the read-capture discipline permits greater concurrency than does mark-point, the correctness argument is a bit more involved. The induction part of the argument goes as follows:

1. The WAIT for PENDING values in READ_CURRENT_VALUE ensures that if any pending transaction $k < n$ has modified any value that is later read by transaction n, transaction n will wait for transaction k to commit or abort.
2. The setting of the high-water mark when transaction n calls READ_CURRENT_VALUE, together with the test of the high-water mark in NEW_VERSION ensures that if any transaction $j < n$ tries to modify any value after transaction n has read that value, transaction j will abort and not modify that value.
3. Therefore, every value that READ_CURRENT_VALUE returns to transaction n will include the final effect of all preceding transactions $1...n-1$.
4. Therefore, every transaction n will act as if it serially follows transaction $n-1$.

Optimistic coordination disciplines such as read-capture have the possibly surprising effect that something done by a transaction later in the serial ordering can cause a transaction earlier in the ordering to abort. This effect is the price of optimism; to be a good candidate for an optimistic discipline, an application probably should not have a lot of data interference.

A subtlety of read-capture is that it is necessary to implement bootstrapping before-or-after atomicity in the procedure NEW_VERSION, by adding a lock and calls to ACQUIRE and RELEASE because NEW_VERSION can now be called by two concurrent threads that happen to add new versions to the same variable at about the same time. In addition, NEW_VERSION must be careful to keep versions of the same variable in transaction order, so that the backward search performed by READ_CURRENT_VALUE works correctly.

There is one final detail, an interaction with all-or-nothing recovery. High water marks should be stored in volatile memory, so that following a crash (which has the effect

我们移除了对并发事务各组成步骤实时顺序的所有约束，因此可能出现高序号事务先创建某对象的新版本，随后低序号事务又试图创建同一对象新版本的情况。由于我们的NEW_VERSION流程只是简单地将新版本附加到对象历史记录的末尾，最终可能导致历史记录顺序错乱。避免这一错误的最简单方法是在NEW_VERSION (的第15)行加入额外检查，确保每个新版本的客户端序列号都大于前一版本的序列号。若不满足此条件，NEW_VERSION将中止该事务，如同发生读捕获冲突时一样。（此检查仅中止执行冲突性*blind writes*的事务，这类情况并不常见。若冲突事务中任一方在写入前先读取该值，通过*high_water_mark* 的设置与检测即可捕获并阻止冲突。）

对于这类算法，人们首先必须提出的问题是它是否真的有效：其结果是否总是与并发事务的某种串行顺序一致？由于读捕获机制允许比标记点更高的并发度，其正确性论证就略显复杂。论证的归纳部分如下：

1. 对于READ_CURRENT_VALUE中的PENDING值，WAIT确保如果有任何待处理事务 $k < n$修改了随后被事务n, 读取的值，事务n将等待事务k提交或中止。 2. 当事务n调用READ_CURRENT_VALUE时设置高水位标记，以及在NEW_VERSION 中对高水位标记的测试，共同保证了如果有任何事务$j < n$ 试图在事务n读取某值后修改该值，事务j将中止而不会修改该值。 3. 因此，READ_CURRENT_VALUE返回给事务n的每一个值都将包含前驱事务$1 \ldots n-1$的最终效果。 4. 因此，每个事务n的行为都将如同它串行跟随事务$n-1$之后一样。

乐观协调机制（如读捕获）可能会带来一个看似矛盾的效果：在串行顺序中较后发生的事务操作，可能导致顺序中较早的事务中止。这种效应是乐观策略的代价所在；要成为乐观机制的理想候选方案，应用场景中最好不存在大量的数据冲突。

读捕获的一个微妙之处在于，必须在过程NEW_VERSION中实现原子性前后的引导，通过添加锁以及对ACQUIRE和RELEASE的调用来完成，因为NEW_VERSION现在可能被两个并发线程调用，这些线程恰好在几乎同一时间为同一变量添加新版本。此外，NEW_VERSION 必须谨慎保持同一变量的版本按照事务顺序排列，以确保READ_CURRENT_VALUE执行的后向搜索能够正确工作。

还有一个最后的细节，涉及到全有或全无恢复的交互。高水位标记应存储在易失性内存中，以便在发生崩溃（其效果

of aborting all pending transactions) the high water marks automatically disappear and thus don't cause unnecessary aborts.

9.4.4 Does Anyone Actually Use Version Histories for Before-or-After Atomicity?

The answer is yes, but the most common use is in an application not likely to be encountered by a software specialist. Legacy processor architectures typically provide a limited number of registers (the "architectural registers") in which the programmer can hold temporary results, but modern large scale integration technology allows space on a physical chip for many more physical registers than the architecture calls for. More registers generally allow better performance, especially in multiple-issue processor designs, which execute several sequential instructions concurrently whenever possible. To allow use of the many physical registers, a register mapping scheme known as *register renaming* implements a version history for the architectural registers. This version history allows instructions that would interfere with each other only because of a shortage of registers to execute concurrently.

For example, Intel Pentium processors, which are based on the x86 instruction set architecture described in Section 5.7, have only eight architectural registers. The Pentium 4 has 128 physical registers, and a register renaming scheme based on a circular *reorder buffer*. A reorder buffer resembles a direct hardware implementation of the procedures NEW_VERSION and WRITE_VALUE of Figure 9.29. As each instruction issues (which corresponds to BEGIN_TRANSACTION), it is assigned the next sequential slot in the reorder buffer. The slot is a map that maintains a correspondence between two numbers: the number of the architectural register that the programmer specified to hold the output value of the instruction, and the number of one of the 128 physical registers, the one that will actually hold that output value. Since machine instructions have just one output value, assigning a slot in the reorder buffer implements in a single step the effect of both NEW_OUTCOME_RECORD and NEW_VERSION. Similarly, when the instruction commits, it places its output in that physical register, thereby implementing WRITE_VALUE and COMMIT as a single step.

Figure 9.34 illustrates register renaming with a reorder buffer. In the program sequence of that example, instruction n uses architectural register five to hold an output value that instruction $n + 1$ will use as an input. Instruction $n + 2$ loads architectural register five from memory. Register renaming allows there to be two (or more) versions of register five simultaneously, one version (in physical register 42) containing a value for use by instructions n and $n + 1$ and the second version (in physical register 29) to be used by instruction $n + 2$. The performance benefit is that instruction $n + 2$ (and any later instructions that write into architectural register 5) can proceed concurrently with instructions n and $n + 1$. An instruction following instruction $n + 2$ that requires the new value in architectural register five as an input uses a hardware implementation of READ_CURRENT_VALUE to locate the most recent preceding mapping of architectural register five in the reorder buffer. In this case that most recent mapping is to physical register 29.

（中止所有待处理事务时）高水位标记会自动消失，因此不会引发不必要的中止。

9.4.4 有人真的使用版本历史来实现前后原子性吗?

答案是肯定的，但最常见的应用场景不太可能被软件专家遇到。传统处理器架构通常只提供有限数量的寄存器（即“架构寄存器”），供程序员暂存临时结果，而现代大规模集成电路技术允许在物理芯片上实现比架构要求多得多的物理寄存器。更多的寄存器通常能带来更好的性能，尤其是在多发射处理器设计中——这类设计会尽可能并行执行多条顺序指令。为了利用这些额外的物理寄存器，一种名为*register renaming*的寄存器映射方案为架构寄存器实现了版本历史记录。这种版本历史使得那些原本仅因寄存器数量不足而相互干扰的指令能够并发执行。

例如，基于第5.7节描述的x86指令集架构的英特尔奔腾处理器仅有八个架构寄存器。而奔腾4则拥有128个物理寄存器，并采用了一种基于环形*reorder buffer*的寄存器重命名方案。重排序缓冲区类似于对图9.29中NEW_VERSION 和 WRITE_VALUE 过程的直接硬件实现。当每条指令发射时（对应BEGIN_TRANSACTION），它会被分配到重排序缓冲区中的下一个顺序槽位。该槽位是一个映射表，维护着两个编号之间的对应关系：程序员指定用于存放指令输出值的架构寄存器编号，以及实际将存储该输出值的128个物理寄存器之一编号。由于机器指令仅有一个输出值，在重排序缓冲区中分配槽位便以单步操作同时实现了NEW_OUTCOME_RECORD和NEW_VERSION的效果。类似地，当指令提交时，它会将其输出值存入对应的物理寄存器，从而以单步操作完成WRITE_VALUE 和COMMIT 的实现。

图9.34展示了采用重排序缓冲区的寄存器重命名机制。在该示例的程序序列中，指令n使用架构寄存器五来保存输出值，该值将被指令$n + 1$作为输入使用。指令$n + 2$从内存加载架构寄存器五的内容。寄存器重命名技术允许同时存在两个（或多个）版本的寄存器五，其中一个版本（位于物理寄存器42中）包含供指令n和$n + 1$使用的数值，而第二个版本（位于物理寄存器29中）则供指令$n + 2$使用。其性能优势在于，指令$n + 2$（以及后续任何向架构寄存器五写入的指令）能够与指令n和$n + 1$并发执行。在指令$n + 2$之后需要以架构寄存器五的新值作为输入的指令，将通过硬件实现的READ_CURRENT_VALUE来定位重排序缓冲区中该架构寄存器最近一次的前置映射。在此例中，该最近映射指向物理寄存器29。

The later instruction then stalls, waiting for instruction *n* + 2 to write a value into physical register 29. Later instructions that reuse architectural register five for some purpose that does not require that version can proceed concurrently.

Although register renaming is conceptually straightforward, the mechanisms that prevent interference when there are dependencies between instructions tend to be more intricate than either of the mark-point or read-capture disciplines, so this description has been oversimplified. For more detail, the reader should consult a textbook on processor architecture, for example *Computer Architecture, a Quantitative Approach*, by Hennessy and Patterson [Suggestions for Further Reading 1.1.1].

The Oracle database management system offers several before-or-after atomicity methods, one of which it calls "serializable", though the label may be a bit misleading. This method uses a before-or-after atomicity scheme that the database literature calls *snapshot isolation*. The idea is that when a transaction begins the system conceptually takes a snapshot of every committed value and the transaction reads all of its inputs from that snapshot. If two concurrent transactions (which might start with the same snapshot) modify the same variable, the first one to commit wins; the system aborts the other one with a "serialization error". This scheme effectively creates a limited variant of a version

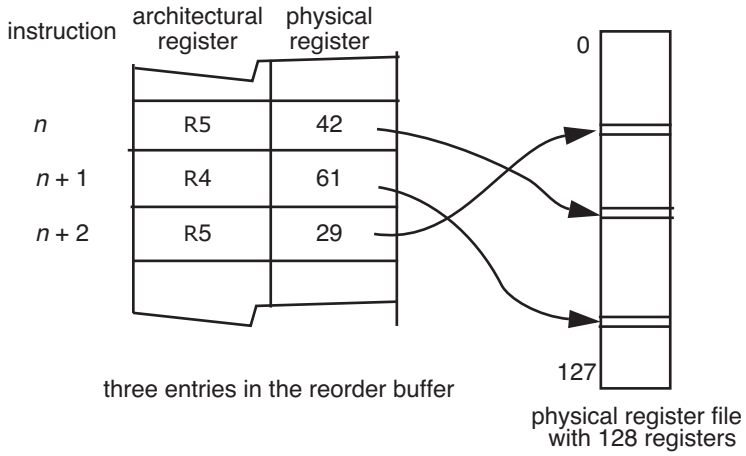

FIGURE 9.34

Example showing how a reorder buffer maps architectural register numbers to physical register numbers. The program sequence corresponding to the three entries is:

```
n       R5 ← R4 × R2            // Write a result in register five.
n + 1   R4 ← R5 + R1            // Use result in register five.
n + 2   R5 ← READ (117492)      // Write content of a memory cell in register five.
```

Instructions *n* and *n* + 2 both write into register R5, so R5 has two versions, with mappings to physical registers 42 and 29, respectively. Instruction *n* + 2 can thus execute concurrently with instructions *n* and *n* + 1.

随后指令会暂停，等待指令$n + 2$将值写入物理寄存器29。而那些重用架构寄存器五用于某些不需要该版本的后续指令则可以并发执行。

尽管寄存器重命名在概念上简单直接，但当指令间存在依赖关系时，防止干扰的机制往往比标记点或读取捕获这两种规则更为复杂，因此上述描述已作了过度简化。欲了解更多细节，读者应参考处理器架构的教科书，例如Hennessy和Patterson所著的*Computer Architecture, a Quantitative Approach*（进一步阅读建议1.1.1）。

Oracle数据库管理系统提供了多种前后原子性方法，其中一种被称为“可序列化”，尽管这一标签可能有些误导性。该方法采用了一种数据库文献中称为*snapshot isolation*的前后原子性方案。其核心思想是，当一个事务开始时，系统会在概念上对所有已提交的值进行一次快照，事务从该快照中读取所有输入值。如果两个并发事务（可能基于同一快照启动）修改了同一变量，先提交的事务将成功；系统会以“序列化错误”为由中止另一事务。此方案实质上创建了一个有限版本的

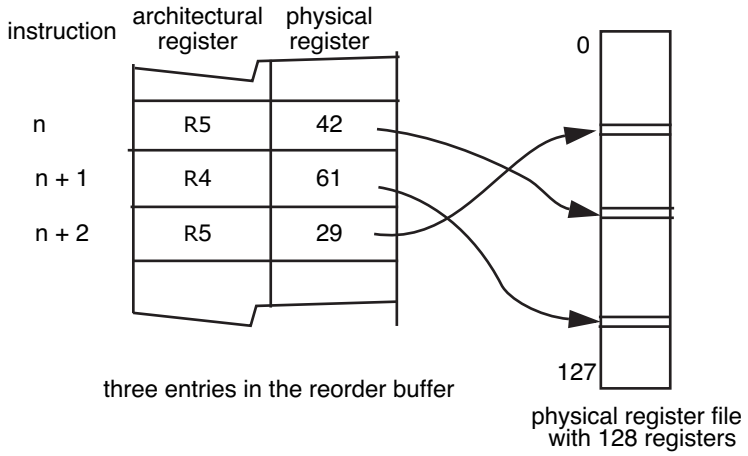

图9.34

示例展示重排序缓冲区如何将架构寄存器编号映射到物理寄存器编号。对应这三个条目的程序序列是：

n R5 ← R4 × R2 // 将结果写入寄存器五。*n* + 1 R4 ← R5 + R1 // 使用寄存器五中的结果。*n* + 2 R5 ← 读取(117492) // 将内存单元的内容写入寄存器五。

指令n和n+ 2都写入寄存器R5，因此R5有两个版本，分别映射到物理寄存器42和29。这样，指令n+ 2就可以与指令n和n+ 1并发执行。

history that, in certain situations, does not always ensure that concurrent transactions are correctly coordinated.

Another specialized variant implementation of version histories, known as *transactional memory*, is a discipline for creating atomic actions from arbitrary instruction sequences that make multiple references to primary memory. Transactional memory was first suggested in 1993 and with widespread availability of multicore processors, has become the subject of quite a bit of recent research interest because it allows the application programmer to use concurrent threads without having to deal with locks. The discipline is to mark the beginning of an instruction sequence that is to be atomic with a "begin transaction" instruction, direct all ensuing STORE instructions to a hidden copy of the data that concurrent threads cannot read, and at end of the sequence check to see that nothing read or written during the sequence was modified by some other transaction that committed first. If the check finds no such earlier modifications, the system commits the transaction by exposing the hidden copies to concurrent threads; otherwise it discards the hidden copies and the transaction aborts. Because it defers all discovery of interference to the commit point this discipline is even more optimistic than the read-capture discipline described in Section 9.4.3 above, so it is most useful in situations where interference between concurrent threads is possible but unlikely. Transactional memory has been experimentally implemented in both hardware and software. Hardware implementations typically involve tinkering with either a cache or a reorder buffer to make it defer writing hidden copies back to primary memory until commit time, while software implementations create hidden copies of changed variables somewhere else in primary memory. As with instruction renaming, this description of transactional memory is somewhat oversimplified, and the interested reader should consult the literature for fuller explanations.

Other software implementations of version histories for before-or-after atomicity have been explored primarily in research environments. Designers of database systems usually use locks rather than version histories because there is more experience in achieving high performance with locks. Before-or-after atomicity by using locks systematically is the subject of the next section of this chapter.

9.5 Before-or-After Atomicity II: Pragmatics

The previous section showed that a version history system that provides all-or-nothing atomicity can be extended to also provide before-or-after atomicity. When the all-or-nothing atomicity design uses a log and installs data updates in cell storage, other, concurrent actions can again immediately see those updates, so we again need a scheme to provide before-or-after atomicity. When a system uses logs for all-or-nothing atomicity, it usually adopts the mechanism introduced in Chapter 5—*locks*—for before-or-after atomicity. However, as Chapter 5 pointed out, programming with locks is hazardous, and the traditional programming technique of debugging until the answers seem to be correct is unlikely to catch all locking errors. We now revisit locks, this time with the goal

在某些情况下，历史并不总能确保并发事务得到正确协调。

版本历史的另一种专门变体实现，称为*transac- tional memory*，是一种从任意指令序列创建原子操作的规范，这些指令序列多次引用主内存。事务内存的概念最早于1993年被提出，随着多核处理器的广泛普及，因其允许应用程序员使用并发线程而无需处理锁的问题，已成为近期不少研究的关注焦点。该规范要求在原子性指令序列起始处标记“开始事务”指令，将所有后续STORE指令导向并发线程无法读取的数据隐藏副本，并在序列结束时检查确认该序列期间读取或写入的内容未被其他已提交事务修改。若检查未发现此类早期修改，系统则通过向并发线程公开隐藏副本来提交事务；否则丢弃隐藏副本并中止事务。由于它将所有冲突发现推迟至提交点，这一规范比上文第9.4.3节所述的读取捕获规范更为乐观，因此最适用于并发线程间可能但不太可能发生冲突的场景。事务内存已在硬件和软件层面得到实验性实现：硬件实现通常涉及调整缓存或重排序缓冲区，使其延迟将隐藏副本写回主内存直至提交时刻；而软件实现则在主内存其他位置创建变更变量的隐藏副本。与指令重命名类似，此处对事务内存的描述略显简化，感兴趣的读者应查阅文献以获取更完整的解释。

其他软件实现版本历史以实现前后原子性的研究主要在科研环境中进行探索。数据库系统的设计者通常倾向于使用锁而非版本历史，因为在使用锁实现高性能方面积累了更多经验。下一节将系统性地探讨通过锁机制实现前后原子性的主题。

9.5 前后原子性II：实用考量

上一节表明，提供全有或全无原子性的版本历史系统可扩展以支持先写后读原子性。当全有或全无原子性设计采用日志并在单元存储中安装数据更新时，其他并发操作能立即看到这些更新，因此我们仍需一种方案来确保先写后读原子性。若系统利用日志实现全有或全无原子性，通常也会采用第5章—*locks*—介绍的机制来实现先写后读原子性。然而，如第5章所述，使用锁进行编程存在风险，依赖传统调试技术直至结果看似正确的方法难以捕捉所有锁错误。现在我们重新审视锁机制，此次目标是

of using them in stylized ways that allow us to develop arguments that the locks correctly implement before-or-after atomicity.

9.5.1 Locks

To review, a *lock* is a flag associated with a data object and set by an action to warn other, concurrent, actions not to read or write the object. Conventionally, a locking scheme involves two procedures:

ACQUIRE (*A.lock*)

marks a lock variable associated with object *A* as having been acquired. If the object is already acquired, ACQUIRE waits until the previous acquirer releases it.

RELEASE (*A.lock*)

unmarks the lock variable associated with *A*, perhaps ending some other action's wait for that lock. For the moment, we assume that the semantics of a lock follow the single-acquire protocol of Chapter 5: if two or more actions attempt to acquire a lock at about the same time, only one will succeed; the others must find the lock already acquired. In Section 9.5.4 we will consider some alternative protocols, for example one that permits several readers of a variable as long as there is no one writing it.

The biggest problem with locks is that programming errors can create actions that do not have the intended before-or-after property. Such errors can open the door to races that, because the interfering actions are timing dependent, can make it extremely difficult to figure out what went wrong. Thus a primary goal is that coordination of concurrent transactions should be arguably correct. For locks, the way to achieve this goal is to follow three steps systematically:

- Develop a locking discipline that specifies which locks must be acquired and when.
- Establish a compelling line of reasoning that concurrent transactions that follow the discipline will have the before-or-after property.
- Interpose a *lock manager*, a program that enforces the discipline, between the programmer and the ACQUIRE and RELEASE procedures.

Many locking disciplines have been designed and deployed, including some that fail to correctly coordinate transactions (for an example, see exercise 9.5). We examine three disciplines that succeed. Each allows more concurrency than its predecessor, though even the best one is not capable of guaranteeing that concurrency is maximized.

The first, and simplest, discipline that coordinates transactions correctly is the *system-wide lock*. When the system first starts operation, it creates a single lockable variable named, for example, *System*, in volatile memory. The discipline is that every transaction must start with

以程式化的方式使用它们，使我们能够论证这些锁正确地实现了前后原子性。

9.5.1 锁

回顾一下，*lock*是与数据对象关联的一个标志，由某个操作设置，用于警告其他并发操作不要读取或写入该对象。传统上，锁定方案涉及两个过程：

获取（{v*}）

标记与对象*A*关联的锁变量为已获取状态。若该对象已被获取，ACQUIRE将等待直至前一个获取者释放它。

发布 (*A.lock*)

解除与*A*关联的锁变量标记，可能结束其他某个动作对该锁的等待。目前，我们假设锁的语义遵循第5章的单次获取协议：若两个或更多动作几乎同时尝试获取锁，只有一个会成功；其他动作将发现锁已被占用。在9.5.4节中，我们将探讨一些替代协议，例如允许变量在无人写入时被多个读取者同时访问的协议。

锁的最大问题在于，编程错误可能导致操作无法实现预期的前后一致性属性。这类错误会为竞态条件敞开大门——由于干扰操作具有时序依赖性，使得查明问题根源变得极其困难。因此，一个核心目标是确保并发事务的协调机制必须无可争议地正确。就锁而言，系统性地遵循三个步骤即可达成这一目标：

- 制定一个锁定规则，明确规定必须获取哪些锁以及何时获取。
- 建立一个令人信服的推理线索，表明遵循该规则的并发事务将具有{v*}之前或之后{v*}的属性。
- 在程序员与ACQUIRE和RELEASE过程之间，插入一个强制执行该规范的*lock manager*程序。

许多锁定机制已被设计和部署，其中一些未能正确协调事务（例如，参见练习9.5）。我们研究了三种成功的机制。每一种都比前一种允许更高的并发性，但即便是最优的方案也无法确保并发性达到最大化。

第一种也是最简单的、能正确协调事务的机制是*system-wide lock*。当系统首次启动运行时，它会在易失性存储器中创建一个可锁变量，例如命名为*System*。该机制的规则是：每笔事务都必须以

```
begin_transaction
ACQUIRE (System.lock)
...
```

and every transaction must end with

```
...
RELEASE (System.lock)
end_transaction
```

A system can even enforce this discipline by including the ACQUIRE and RELEASE steps in the code sequence generated for **begin_transaction** and **end_transaction**, independent of whether the result was COMMIT or ABORT. Any programmer who creates a new transaction then has a guarantee that it will run either before or after any other transactions.

The systemwide lock discipline allows only one transaction to execute at a time. It serializes potentially concurrent transactions in the order that they call ACQUIRE. The systemwide lock discipline is in all respects identical to the simple serialization discipline of Section 9.4. In fact, the simple serialization pseudocode

```
id ← NEW_OUTCOME_RECORD ()
preceding_id ← id - 1
wait until preceding_id.outcome_record.value ≠ PENDING
...
COMMIT (id) [or ABORT (id)]
```

and the systemwide lock invocation

```
ACQUIRE (System.lock)
...
RELEASE (System.lock)
```

are actually just two implementations of the same idea.

As with simple serialization, systemwide locking restricts concurrency in cases where it doesn't need to because it locks all data touched by every transaction. For example, if systemwide locking were applied to the funds TRANSFER program of Figure 9.16, only one transfer could occur at a time, even though any individual transfer involves only two out of perhaps several million accounts, so there would be many opportunities for concurrent, non-interfering transfers. Thus there is an interest in developing less restrictive locking disciplines. The starting point is usually to employ a finer lock *granularity*: lock smaller objects, such as individual data records, individual pages of data records, or even fields within records. The trade-offs in gaining concurrency are first, that when there is more than one lock, more time is spent acquiring and releasing locks and second, correctness arguments become more complex. One hopes that the performance gain from concurrency exceeds the cost of acquiring and releasing the multiple locks. Fortunately, there are at least two other disciplines for which correctness arguments are feasible, *simple locking* and *two-phase locking*.

```
开始事务 获取(
System.lock) …
```

且每笔交易必须以

```
… 释放 (System.lock)
结束事务
```

一个系统甚至可以通过在**begin_transaction** 和 **end_transaction**生成的代码序列中包含ACQUIRE和RELEASE步骤来强制执行这一规则，无论结果是COMMIT还是ABORT。这样，任何创建新事务的程序员都能确保该事务将在其他任何事务之前或之后运行。

系统范围的锁机制只允许一个事务在同一时间执行。它按照事务调用ACQUIRE的顺序，将潜在的并发事务串行化。系统范围的锁机制在所有方面都与第9.4节中简单的串行化机制相同。实际上，简单的串行化伪代码

```
id ← 新结果记录 () preceding_id ← id - 1 等待直到
preceding_id.outcome_record.value ≠ 待处理

…… 提交（id） [或 中止（id
） ]
```

以及系统范围的锁调用

```
获取 (System.lock)
… 发布 (System.lock)
```

实际上只是同一想法的两种实现方式。

与简单串行化一样，系统级锁定在无需限制并发的情况下也施加了约束，因为它会锁定每笔事务涉及的所有数据。例如，若将系统级锁定应用于图9.16中的资金TRANSFER程序，同一时间只能执行一次转账操作——即便单次转账仅涉及数百万账户中的两个账户，这显然错失了大量可并行执行且互不干扰的转账机会。因此，开发限制更少的锁定机制成为研究重点。通常的切入点是采用更细粒度的锁*granularity*：锁定更小的对象单元，如单条数据记录、数据记录的单个页面，乃至记录内的字段。提升并发性需要权衡的是：首先，多锁机制会增加获取与释放锁的时间开销；其次，正确性论证会变得更加复杂。人们期望并发带来的性能提升能超越多锁操作的成本。幸运的是，至少还存在另外两种可实现正确性论证的机制*simple locking*与*two-phase locking*。

9.5.2 Simple Locking

The second locking discipline, known as *simple locking,* is similar in spirit to, though not quite identical with, the mark-point discipline. The simple locking discipline has two rules. First, each transaction must acquire a lock for every shared data object it intends to read or write before doing any actual reading and writing. Second, it may release its locks only after the transaction installs its last update and commits or completely restores the data and aborts. Analogous to the mark point, the transaction has what is called a *lock point*: the first instant at which it has acquired all of its locks. The collection of locks it has acquired when it reaches its lock point is called its *lock set*. A lock manager can enforce simple locking by requiring that each transaction supply its intended lock set as an argument to the **begin_transaction** operation, which acquires all of the locks of the lock set, if necessary waiting for them to become available. The lock manager can also interpose itself on all calls to read data and to log changes, to verify that they refer to variables that are in the lock set. The lock manager also intercepts the call to commit or abort (or, if the application uses roll-forward recovery, to log an END record) at which time it automatically releases all of the locks of the lock set.

The simple locking discipline correctly coordinates concurrent transactions. We can make that claim using a line of argument analogous to the one used for correctness of the mark-point discipline. Imagine that an all-seeing outside observer maintains an ordered list to which it adds each transaction identifier as soon as the transaction reaches its lock point and removes it from the list when it begins to release its locks. Under the simple locking discipline each transaction has agreed not to read or write anything until that transaction has been added to the observer's list. We also know that all transactions that precede this one in the list must have already passed their lock point. Since no data object can appear in the lock sets of two transactions, no data object in any transaction's lock set appears in the lock set of the transaction preceding it in the list, and by induction to any transaction earlier in the list. Thus all of this transaction's input values are the same as they will be when the preceding transaction in the list commits or aborts. The same argument applies to the transaction before the preceding one, so all inputs to any transaction are identical to the inputs that would be available if all the transactions ahead of it in the list ran serially, in the order of the list. Thus the simple locking discipline ensures that this transaction runs completely after the preceding one and completely before the next one. Concurrent transactions will produce results as if they had been serialized in the order that they reached their lock points.

As with the mark-point discipline, simple locking can miss some opportunities for concurrency. In addition, the simple locking discipline creates a problem that can be significant in some applications. Because it requires the transaction to acquire a lock on every shared object that it will either read *or* write (recall that the mark-point discipline requires marking only of shared objects that the transaction will write), applications that discover which objects need to be read by reading other shared data objects have no alternative but to lock every object that they *might* need to read. To the extent that the set of objects that an application *might* need to read is larger than the set for which it eventually

9.5.2 简单锁定

第二种锁定规则，称为*simple locking,*，在精神上与标记点规则相似，但并不完全相同。简单锁定规则包含两条原则：首先，每个事务在真正进行读写操作前，必须为所有打算读取或写入的共享数据对象获取锁；其次，只有在事务安装完最后一次更新并提交，或完全恢复数据并中止后，才能释放其持有的锁。与标记点类似，事务存在一个被称为*lock point*的关键时刻：即事务获取全部所需锁的第一个瞬间。当事务达到其锁定点时已获得的锁集合称为其 *lock set*。锁管理器可通过要求每个事务将其预期锁集作为参数传递给**begin_transaction**操作（该操作会获取锁集中的所有锁，必要时等待锁可用）来强制执行简单锁定。锁管理器还能拦截所有数据读取和变更日志调用，以验证这些操作是否针对锁集内的变量。此外，锁管理器会截获提交或中止调用（若应用程序采用前滚恢复，则截获记录END的日志调用），此时它将自动释放锁集中的所有锁。

简单的锁定规则正确地协调了并发事务。我们可以采用类似于证明标记点规则正确性的论证思路来支持这一主张。设想一位全知的外部观察者维护着一个有序列表，每当一个事务到达其锁定点时，观察者就将该事务标识符加入列表；当事务开始释放锁时，则将其从列表中移除。在简单锁定规则下，每个事务都承诺在自身被加入观察者列表之前不会读取或写入任何数据。我们还知道，列表中排在该事务之前的所有事务必定已经通过了它们的锁定点。由于任何数据对象都不可能同时出现在两个事务的锁定集合中，因此任一事务锁定集合中的数据对象都不会出现在列表中前驱事务的锁定集合中，通过归纳推理可知也不会出现在更早事务的锁定集合中。因此，该事务的所有输入值都与列表中前驱事务提交或中止时的状态保持一致。同样的论证适用于前驱事务的前序事务，因此任何事务的输入都与假设列表中所有前序事务按列表顺序串行执行时所能提供的输入完全相同。由此可见，简单锁定规则确保了当前事务完全在前驱事务之后执行，并完全在后续事务之前执行。并发事务产生的结果，将如同它们按照到达锁定点的顺序被串行化执行一样。

与标记点规则一样，简单的锁定机制可能会错过一些并发机会。此外，这种简单的锁定规则还会带来一个问题，在某些应用中可能相当显著。因为它要求事务对每一个将要读取*or*或写入的共享对象获取锁（回想一下，标记点规则仅要求对事务将要写入的共享对象进行标记），对于那些需要通过读取其他共享数据对象来发现需要读取哪些对象的应用来说，除了锁定所有*might*需要读取的对象外别无选择。就应用*might*需要读取的对象集合最终大于其实际操作的集合而言

does read, the simple locking discipline can interfere with opportunities for concurrency. On the other hand, when the transaction is straightforward (such as the TRANSFER transaction of Figure 9.16, which needs to lock only two records, both of which are known at the outset) simple locking can be effective.

9.5.3 Two-Phase Locking

The third locking discipline, called *two-phase locking*, like the read-capture discipline, avoids the requirement that a transaction must know in advance which locks to acquire. Two-phase locking is widely used, but it is harder to argue that it is correct. The two-phase locking discipline allows a transaction to acquire locks as it proceeds, and the transaction may read or write a data object as soon as it acquires a lock on that object. The primary constraint is that the transaction may not release any locks until it passes its lock point. Further, the transaction can release a lock on an object that it only reads any time after it reaches its lock point *if* it will never need to read that object again, even to abort. The name of the discipline comes about because the number of locks acquired by a transaction monotonically increases up to the lock point (the first phase), after which it monotonically decreases (the second phase). Just as with simple locking, two-phase locking orders concurrent transactions so that they produce results as if they had been serialized in the order they reach their lock points. A lock manager can implement two-phase locking by intercepting all calls to read and write data; it acquires a lock (perhaps having to wait) on the first use of each shared variable. As with simple locking, it then holds the locks until it intercepts the call to commit, abort, or log the END record of the transaction, at which time it releases them all at once.

The extra flexibility of two-phase locking makes it harder to argue that it guarantees before-or-after atomicity. Informally, once a transaction has acquired a lock on a data object, the value of that object is the same as it will be when the transaction reaches its lock point, so reading that value now must yield the same result as waiting till then to read it. Furthermore, releasing a lock on an object that it hasn't modified must be harmless if this transaction will never look at the object again, even to abort. A formal argument that two-phase locking leads to correct before-or-after atomicity can be found in most advanced texts on concurrency control and transactions. See, for example, *Transaction Processing*, by Gray and Reuter [Suggestions for Further Reading 1.1.5].

The two-phase locking discipline can potentially allow more concurrency than the simple locking discipline, but it still unnecessarily blocks certain serializable, and therefore correct, action orderings. For example, suppose transaction T_1 reads *X* and writes *Y*, while transaction T_2 just does a (blind) write to *Y*. Because the lock sets of T_1 and T_2 intersect at variable *Y*, the two-phase locking discipline will force transaction T_2 to run either completely before or completely after T_1. But the sequence

T_1: READ *X*
T_2: WRITE *Y*
T_1: WRITE *Y*

*does*读取时，简单的锁定机制可能会干扰并发机会。另一方面，当事务较为直接（例如图9.16中的TRANSFER事务，它只需锁定两条记录，且这两条记录在开始时即已知晓）时，简单的锁定机制可以很有效。

9.5.3 两阶段锁定

第三种锁定规则，与读捕获规则一样被称为*two-phase locking,*，避免了事务必须预先知道要获取哪些锁的要求。两阶段锁定被广泛使用，但更难论证其正确性。两阶段锁定规则允许事务在执行过程中获取锁，一旦事务获取了某个数据对象的锁，就可以立即读取或写入该对象。主要约束是，在事务通过其锁定点之前，不得释放任何锁。此外，对于仅读取的对象，事务在达到锁定点*if*后可以随时释放其锁——即使是为了中止，也无需再次读取该对象。 该规则的名称源于事务获取的锁数量在达到锁定点之前单调递增（第一阶段），之后则单调递减（第二阶段）。与简单锁定一样，两阶段锁定通过按事务达到锁定点的顺序对并发事务进行排序，使它们产生的结果如同被串行化执行。锁管理器可以通过拦截所有读写数据的调用来实现两阶段锁定；它在首次使用每个共享变量时获取锁（可能需要等待）。与简单锁定类似，锁管理器会一直持有这些锁，直到拦截到提交、中止或记录事务END记录的调用，此时它会一次性释放所有锁。

两阶段锁的额外灵活性使得更难论证它能确保前后原子性。非正式地说，一旦事务获取了数据对象上的锁，该对象的值将与事务到达其锁定点时相同，因此现在读取该值必然与等到那时再读取得到相同结果。此外，如果事务之后不再查看该对象（即使是回滚时），释放未修改对象上的锁必定无害。关于两阶段锁能实现正确前后原子性的形式化论证，可在多数关于并发控制与事务处理的高级文献中找到，例如Gray和Reuter所著的*Trans-action Processing*（参见[进一步阅读建议1.1.5]）。

两阶段锁定协议相比简单锁定协议可能允许更高的并发度，但它仍然会不必要地阻塞某些可串行化（因而正确）的操作序列。例如，假设事务T1读取*x*并写入*y*，而事务T2仅对*y*执行（盲目）写入。由于T1和T2的锁集在变量*y*处相交，两阶段锁定协议将强制事务T2要么完全在T1之前运行，要么完全在T1之后运行。但序列

T1: 读取 *X*
T2: 写入 *Y* T
1: 写入 *Y*

in which the write of T_2 occurs between the two steps of T_1, yields the same result as running T_2 completely before T_1, so the result is always correct, even though this sequence would be prevented by two-phase locking. Disciplines that allow all possible concurrency while at the same time ensuring before-or-after atomicity are quite difficult to devise. (Theorists identify the problem as NP-complete.)

There are two interactions between locks and logs that require some thought: (1) individual transactions that abort, and (2) system recovery. Aborts are the easiest to deal with. Since we require that an aborting transaction restore its changed data objects to their original values before releasing any locks, no special account need be taken of aborted transactions. For purposes of before-or-after atomicity they look just like committed transactions that didn't change anything. The rule about not releasing any locks on modified data before the end of the transaction is essential to accomplishing an abort. If a lock on some modified object were released, and then the transaction decided to abort, it might find that some other transaction has now acquired that lock and changed the object again. Backing out an aborted change is likely to be impossible unless the locks on modified objects have been held.

The interaction between log-based recovery and locks is less obvious. The question is whether locks themselves are data objects for which changes should be logged. To analyze this question, suppose there is a system crash. At the completion of crash recovery there should be no pending transactions because any transactions that were pending at the time of the crash should have been rolled back by the recovery procedure, and recovery does not allow any new transactions to begin until it completes. Since locks exist only to coordinate pending transactions, it would clearly be an error if there were locks still set when crash recovery is complete. That observation suggests that locks belong in volatile storage, where they will automatically disappear on a crash, rather than in nonvolatile storage, where the recovery procedure would have to hunt them down to release them. The bigger question, however, is whether or not the log-based recovery algorithm will construct a correct system state—correct in the sense that it could have arisen from some serial ordering of those transactions that committed before the crash.

Continue to assume that the locks are in volatile memory, and at the instant of a crash all record of the locks is lost. Some set of transactions—the ones that logged a BEGIN record but have not yet logged an END record—may not have been completed. But we know that the transactions that were not complete at the instant of the crash had nonoverlapping lock sets at the moment that the lock values vanished. The recovery algorithm of Figure 9.23 will systematically UNDO or REDO installs for the incomplete transactions, but every such UNDO or REDO must modify a variable whose lock was in some transaction's lock set at the time of the crash. Because those lock sets must have been non-overlapping, those particular actions can safely be redone or undone without concern for before-or-after atomicity during recovery. Put another way, the locks created a particular serialization of the transactions and the log has captured that serialization. Since RECOVER performs UNDO actions in reverse order as specified in the log, and it performs REDO actions in forward order, again as specified in the log, RECOVER reconstructs exactly that same serialization. Thus even a recovery algorithm that reconstructs the

其中，T2的写入操作发生在T1的两个步骤之间，产生的结果与在T1之前完全运行T2相同，因此结果始终是正确的，尽管这种顺序会被两阶段锁定所阻止。要设计出既能允许所有可能的并发，又能确保前后原子性的规则是相当困难的。（理论家们将这一问题识别为NP完全问题。）

锁与日志之间存在两种需要仔细考虑的交互情况：(1) 单个事务的中止，(2) 系统恢复。其中中止事务最容易处理。由于我们要求中止事务在释放任何锁之前，必须将其修改的数据对象恢复为原始值，因此无需对中止事务进行特殊处理。就"全有或全无"原子性而言，它们看起来就像未更改任何内容的已提交事务。关于在事务结束前不释放任何已修改数据上的锁这一规则，是实现中止操作的关键。如果某个已修改对象上的锁被释放，而后事务决定中止，它可能会发现其他事务已获取该锁并再次更改了对象。除非持有已修改对象上的锁，否则回滚中止的更改很可能是无法实现的。

基于日志的恢复与锁之间的交互关系不那么显而易见。问题在于锁本身是否应被视为需要记录变更的数据对象。为分析这一问题，假设系统发生崩溃。在崩溃恢复完成后，不应存在任何挂起事务——因为在崩溃时处于挂起状态的所有事务都应由恢复程序回滚，且恢复过程完成前不允许启动新事务。由于锁的存在仅用于协调挂起事务，若崩溃恢复完成后仍有锁被持有，显然属于错误状态。这一观察表明，锁应存放在易失性存储中（崩溃时自动消失），而非非易失性存储（恢复程序需主动追踪并释放它们）。但更核心的问题是：基于日志的恢复算法能否构建出正确的系统状态——这里的"正确"指该状态可能由崩溃前已提交事务的某种串行排序所产生。

继续假设锁位于易失性内存中，在崩溃瞬间所有锁记录都会丢失。某些事务集合——那些已记录BEGIN但尚未记录END的事务——可能尚未完成。但我们知道，在崩溃瞬间未完成的事务，在锁值消失时其锁集合是不重叠的。图9.23的恢复算法会系统性地为未完成事务执行UNDO或REDO安装，但每个这样的UNDO或REDO操作必须修改一个在崩溃时属于某事务锁集合的变量。由于这些锁集合必然互不重叠，这些特定操作可以安全地重做或撤销，无需担心恢复期间的先后原子性。换言之，锁确立了事务的一种特定串行顺序，而日志已捕获该顺序。由于RECOVER按日志指定的逆序执行UNDO操作，并按日志指定的正序执行REDO操作，RECOVER便精确重建了相同的串行顺序。因此，即便是重构该顺序的恢复算法...

entire database from the log is guaranteed to produce the same serialization as when the transactions were originally performed. So long as no new transactions begin until recovery is complete, there is no danger of miscoordination, despite the absence of locks during recovery.

9.5.4 Performance Optimizations

Most logging-locking systems are substantially more complex than the description so far might lead one to expect. The complications primarily arise from attempts to gain performance. In Section 9.3.6 we saw how buffering of disk I/O in a volatile memory cache, to allow reading, writing, and computation to go on concurrently, can complicate a logging system. Designers sometimes apply two performance-enhancing complexities to locking systems: physical locking and adding lock compatibility modes.

A performance-enhancing technique driven by buffering of disk I/O and physical media considerations is to choose a particular lock granularity known as *physical locking*. If a transaction makes a change to a six-byte object in the middle of a 1000-byte disk sector, or to a 1500-byte object that occupies parts of two disk sectors, there is a question about which "variable" should be locked: the object, or the disk sector(s)? If two concurrent threads make updates to unrelated data objects that happen to be stored in the same disk sector, then the two disk writes must be coordinated. Choosing the right locking granularity can make a big performance difference.

Locking application-defined objects without consideration of their mapping to physical disk sectors is appealing because it is understandable to the application writer. For that reason, it is usually called *logical locking*. In addition, if the objects are small, it apparently allows more concurrency: if another transaction is interested in a different object that is in the same disk sector, it could proceed in parallel. However, a consequence of logical locking is that logging must also be done on the same logical objects. Different parts of the same disk sector may be modified by different transactions that are running concurrently, and if one transaction commits but the other aborts neither the old nor the new disk sector is the correct one to restore following a crash; the log entries must record the old and new values of the individual data objects that are stored in the sector. Finally, recall that a high-performance logging system with a cache must, at commit time, force the log to disk and keep track of which objects in the cache it is safe to write to disk without violating the write-ahead log protocol. So logical locking with small objects can escalate cache record-keeping.

Backing away from the details, high-performance disk management systems typically require that the argument of a PUT call be a block whose size is commensurate with the size of a disk sector. Thus the real impact of logical locking is to create a layer between the application and the disk management system that presents a logical, rather than a physical, interface to its transaction clients; such things as data object management and garbage collection within disk sectors would go into this layer. The alternative is to tailor the logging and locking design to match the native granularity of the disk management system. Since matching the logging and locking granularity to the disk write granularity

从日志中恢复整个数据库可以确保产生与事务最初执行时相同的序列化结果。只要在恢复完成前不开始新的事务，尽管恢复期间没有锁机制，也不会出现协调错误。

9.5.4 性能优化

大多数日志锁定系统的实际复杂度远超目前描述所可能引发的预期。这种复杂性主要源于对性能提升的追求。在9.3.6节中，我们已看到如何通过易失性内存缓存对磁盘I/O进行缓冲——以允许读取、写入和计算并发进行——这会使日志系统变得复杂。设计者有时会为锁定系统引入两种提升性能的复杂机制：物理锁定和增加锁兼容模式。

一种通过磁盘I/O缓冲和物理介质考量驱动的性能优化技术，是选择一种特定的锁粒度，称为*physical locking*。如果事务对一个位于1000字节磁盘扇区中间的六字节对象进行修改，或对一个占据两个磁盘扇区部分空间的1500字节对象进行修改，就会产生一个问题：应该锁定哪个“变量”——是对象本身，还是磁盘扇区？若两个并发线程对恰好存储在同一磁盘扇区中的无关数据对象进行更新，那么这两个磁盘写入操作必须协调进行。选择合适的锁粒度能显著影响性能表现。

不考虑应用定义的对象与物理磁盘扇区的映射关系而直接锁定这些对象，对应用程序编写者来说易于理解，因此颇具吸引力。正因如此，这种方式通常被称为*logical locking.* 。此外，若对象体积较小，它显然能支持更高的并发度：若另一事务对同一磁盘扇区内的不同对象感兴趣，便可并行执行。然而，逻辑锁定的一个必然结果是，日志记录也必须以相同的逻辑对象为单位进行。同一磁盘扇区的不同部分可能被并发运行的不同事务修改，若一个事务提交而另一个中止，则崩溃后既不能恢复为旧扇区状态也不能采用新扇区状态；日志条目必须记录扇区内各数据对象的旧值和新值。最后需注意，采用缓存的高性能日志系统在提交时，必须强制将日志写入磁盘，并追踪缓存中哪些对象可安全写入磁盘而不违反预写日志协议。因此，对小对象实施逻辑锁定可能加剧缓存记录维护的复杂度。

从细节上退一步来看，高性能磁盘管理系统通常要求PUT调用的参数是一个与磁盘扇区大小相匹配的块。因此，逻辑锁定的真正作用是在应用程序与磁盘管理系统之间创建一层，向其事务客户端提供一个逻辑接口而非物理接口；诸如磁盘扇区内的数据对象管理和垃圾回收等功能将被纳入这一层。另一种选择是调整日志记录和锁定设计，以匹配磁盘管理系统的原生粒度。由于将日志记录和锁定的粒度与磁盘写入粒度对齐

can reduce the number of disk operations, both logging changes to and locking blocks that correspond to disk sectors rather than individual data objects is a common practice.

Another performance refinement appears in most locking systems: the specification of *lock compatibility modes*. The idea is that when a transaction acquires a lock, it can specify what operation (for example, READ or WRITE) it intends to perform on the locked data item. If that operation is compatible—in the sense that the result of concurrent transactions is the same as some serial ordering of those transactions—then this transaction can be allowed to acquire a lock even though some other transaction has already acquired a lock on that same data object.

The most common example involves replacing the single-acquire locking protocol with the *multiple-reader, single-writer protocol*. According to this protocol, one can allow any number of readers to simultaneously acquire read-mode locks for the same object. The purpose of a read-mode lock is to ensure that no other thread can change the data while the lock is held. Since concurrent readers do not present an update threat, it is safe to allow any number of them. If another transaction needs to acquire a write-mode lock for an object on which several threads already hold read-mode locks, that new transaction will have to wait for all of the readers to release their read-mode locks. There are many applications in which a majority of data accesses are for reading, and for those applications the provision of read-mode lock compatibility can reduce the amount of time spent waiting for locks by orders of magnitude. At the same time, the scheme adds complexity, both in the mechanics of locking and also in policy issues, such as what to do if, while a prospective writer is waiting for readers to release their read-mode locks, another thread calls to acquire a read-mode lock. If there is a steady stream of arriving readers, a writer could be delayed indefinitely.

This description of performance optimizations and their complications is merely illustrative, to indicate the range of opportunities and kinds of complexity that they engender; there are many other performance-enhancement techniques, some of which can be effective, and others that are of dubious value; most have different values depending on the application. For example, some locking disciplines compromise before-or-after atomicity by allowing transactions to read data values that are not yet committed. As one might expect, the complexity of reasoning about what can or cannot go wrong in such situations escalates. If a designer intends to implement a system using performance enhancements such as buffering, lock compatibility modes, or compromised before-or-after atomicity, it would be advisable to study carefully the book by Gray and Reuter, as well as existing systems that implement similar enhancements.

9.5.5 Deadlock; Making Progress

Section 5.2.5 of Chapter 5 introduced the emergent problem of *deadlock*, the wait-for graph as a way of analyzing deadlock, and lock ordering as a way of preventing deadlock. With transactions and the ability to undo individual actions or even abort a transaction completely we now have more tools available to deal with deadlock, so it is worth revisiting that discussion.

可以减少磁盘操作次数，将更改记录和锁定块对应于磁盘扇区而非单个数据对象是一种常见做法。

大多数锁定系统中还存在另一种性能优化：即*lock compatibility modes*的设定。其核心思想在于，当一个事务获取锁时，可以声明它打算对被锁数据项执行何种操作（例如READ或WRITE）。若该操作是兼容的——即并发事务的结果等同于这些事务的某种串行顺序执行结果——那么即使其他事务已对同一数据对象持有锁，当前事务仍可获准加锁。

最常见的例子是用*multiple-reader, single-writer protocol*替代单获取锁定协议。根据该协议，允许多个读取者同时为同一对象获取读模式锁。读模式锁的作用是确保在持有锁期间，其他线程无法更改数据。由于并发读取者不会构成更新威胁，因此允许任意数量的读取者是安全的。如果另一个事务需要为已被多个线程持有读模式锁的对象获取写模式锁，该新事务必须等待所有读取者释放其读模式锁。在许多应用中，大部分数据访问都是读取操作，对于这些应用，提供读模式锁兼容性可以将等待锁的时间减少数个数量级。同时，该方案增加了复杂性，既包括锁定机制本身，也涉及策略问题，例如，当一个潜在的写入者正在等待读取者释放读模式锁时，另一个线程请求获取读模式锁该如何处理。如果不断有新的读取者到来，写入者可能会被无限期延迟。

这段关于性能优化及其复杂性的描述仅是示例性的，旨在说明它们所带来的各种机会和复杂类型；还有许多其他性能提升技术，其中一些可能有效，而另一些则价值存疑；大多数技术的价值因应用场景而异。例如，某些锁定机制会通过允许事务读取尚未提交的数据值，从而在"前或后"原子性上做出妥协。正如人们所料，在这种情况下，关于哪些可能出错、哪些不会出错的推理复杂性会急剧上升。如果设计者打算利用诸如缓冲、锁兼容模式或妥协的"前或后"原子性等性能增强技术来实现系统，建议仔细研读Gray和Reuter的著作，以及已实现类似增强的现有系统。

9.5.5 死锁；确保进展

第五章的5.2.5节介绍了*deadlock*这一突发现象，等待图作为分析死锁的一种方法，以及锁排序作为预防死锁的手段。随着事务的出现，以及能够撤销单个操作甚至完全中止事务的能力，我们现在拥有了更多应对死锁的工具，因此值得重新审视这一讨论。

The possibility of deadlock is an inevitable consequence of using locks to coordinate concurrent activities. Any number of concurrent transactions can get hung up in a deadlock, either waiting for one another, or simply waiting for a lock to be released by some transaction that is already deadlocked. Deadlock leaves us a significant loose end: correctness arguments ensure us that any transactions that complete will produce results as though they were run serially, but they say nothing about whether or not any transaction will ever complete. In other words, our system may ensure *correctness*, in the sense that no wrong answers ever come out, but it does not ensure *progress*—no answers may come out at all.

As with methods for concurrency control, methods for coping with deadlock can also be described as pessimistic or optimistic. Pessimistic methods take *a priori* action to prevent deadlocks from happening. Optimistic methods allow concurrent threads to proceed, detect deadlocks if they happen, and then take action to fix things up. Here are some of the most popular methods:

1. *Lock ordering* (pessimistic). As suggested in Chapter 5, number the locks uniquely, and require that transactions acquire locks in ascending numerical order. With this plan, when a transaction encounters an already-acquired lock, it is always safe to wait for it, since the transaction that previously acquired it cannot be waiting for any locks that this transaction has already acquired—all those locks are lower in number than this one. There is thus a guarantee that somewhere, at least one transaction (the one holding the highest-numbered lock) can always make progress. When that transaction finishes, it will release all of its locks, and some other transaction will become the one that is guaranteed to be able to make progress. A generalization of lock ordering that may eliminate some unnecessary waits is to arrange the locks in a lattice and require that they be acquired in some lattice traversal order. The trouble with lock ordering, as with simple locking, is that some applications may not be able to predict all of the locks they need before acquiring the first one.

2. *Backing out* (optimistic): An elegant strategy devised by Andre Bensoussan in 1966 allows a transaction to acquire locks in any order, but if it encounters an already-acquired lock with a number lower than one it has previously acquired itself, the transaction must back up (in terms of this chapter, UNDO previous actions) just far enough to release its higher-numbered locks, wait for the lower-numbered lock to become available, acquire that lock, and then REDO the backed-out actions.

3. *Timer expiration* (optimistic). When a new transaction begins, the lock manager sets an interrupting timer to a value somewhat greater than the time it should take for the transaction to complete. If a transaction gets into a deadlock, its timer will expire, at which point the system aborts that transaction, rolling back its changes and releasing its locks in the hope that the other transactions involved in the deadlock may be able to proceed. If not, another one will time out, releasing further locks. Timing out deadlocks is effective, though it has the usual defect: it

使用锁来协调并发活动不可避免地会导致死锁的可能性。任意数量的并发事务都可能陷入死锁状态，要么相互等待，要么只是等待某个已经陷入死锁的事务释放锁。死锁给我们留下了一个重大的未解问题：正确性论证确保我们，任何完成的事务都会产生如同它们串行运行一样的结果，但它们并未说明任何事务是否最终会完成。换句话说，我们的系统可以确保*correctness*，即永远不会产生错误的答案，但它不能确保*progress*——可能根本不会有任何答案输出。

与并发控制方法类似，处理死锁的方法也可分为悲观或乐观两类。悲观方法采取*a priori*措施预先阻止死锁发生。乐观方法则允许多个线程并发执行，一旦检测到死锁便采取行动进行修复。以下列举了几种最常用的方法：

1. *Lock ordering*(悲观)。如第5章所述，为锁分配唯一编号，并要求事务按数字升序获取锁。采用此方案时，当一个事务遇到已被占用的锁时，等待总是安全的，因为先前获取该锁的事务不可能正在等待当前事务已持有的任何锁——那些锁的编号都比当前锁小。因此可以保证至少有一个事务（持有最大编号锁的事务）总能取得进展。当该事务完成时，它将释放所有锁，此时另一个事务将成为被确保能取得进展的事务。锁排序的一种广义化改进是将其组织为格结构，并要求按某种格遍历顺序获取锁，这可能会消除不必要的等待。但锁排序与简单锁机制存在相同问题：某些应用可能无法在获取第一个锁之前预知所需的所有锁。

2. *Backing out*（乐观策略）： Andre Bensoussan于1966年设计的一种优雅策略允许事务以任意顺序获取锁，但如果遇到一个已被获取且编号低于自身已获取锁的锁时，该事务必须回退（根据本章术语，即UNDO之前的操作）至足以释放其高编号锁的程度，等待低编号锁变为可用，获取该锁后，再REDO重新执行被回退的操作。

3. *Timer expiration*（乐观）。当新事务开始时，锁管理器会设置一个中断计时器，其值略大于该事务预计完成所需的时间。如果事务陷入死锁，其计时器将到期，此时系统会中止该事务，回滚其更改并释放其持有的锁，以期参与死锁的其他事务能够继续执行。若仍无法解决，另一个事务将超时，进而释放更多锁。虽然死锁超时机制行之有效，但它也存在常见缺陷：

is difficult to choose a suitable timer value that keeps things moving along but also accommodates normal delays and variable operation times. If the environment or system load changes, it may be necessary to readjust all such timer values, an activity that can be a real nuisance in a large system.

4. *Cycle detection* (optimistic). Maintain, in the lock manager, a wait-for graph (as described in Section 5.2.5) that shows which transactions have acquired which locks and which transactions are waiting for which locks. Whenever another transaction tries to acquire a lock, finds it is already locked, and proposes to wait, the lock manager examines the graph to see if waiting would produce a cycle, and thus a deadlock. If it would, the lock manager selects some cycle member to be a victim, and unilaterally aborts that transaction, so that the others may continue. The aborted transaction then retries in the hope that the other transactions have made enough progress to be out of the way and another deadlock will not occur.

When a system uses lock ordering, backing out, or cycle detection, it is common to also set a timer as a safety net because a hardware failure or a programming error such as an endless loop can create a progress-blocking situation that none of the deadlock detection methods can catch.

Since a deadlock detection algorithm can introduce an extra reason to abort a transaction, one can envision pathological situations where the algorithm aborts every attempt to perform some particular transaction, no matter how many times its invoker retries. Suppose, for example, that two threads named Alphonse and Gaston get into a deadlock trying to acquire locks for two objects named Apple and Banana: Alphonse acquires the lock for Apple, Gaston acquires the lock for Banana, Alphonse tries to acquire the lock for Banana and waits, then Gaston tries to acquire the lock for Apple and waits, creating the deadlock. Eventually, Alphonse times out and begins rolling back updates in preparation for releasing locks. Meanwhile, Gaston times out and does the same thing. Both restart, and they get into another deadlock, with their timers set to expire exactly as before, so they will probably repeat the sequence forever. Thus we still have no guarantee of progress. This is the emergent property that Chapter 5 called *livelock*, since formally no deadlock ever occurs and both threads are busy doing something that looks superficially useful.

One way to deal with livelock is to apply a randomized version of a technique familiar from Chapter 7[on-line]: *exponential random backoff*. When a timer expiration leads to an abort, the lock manager, after clearing the locks, delays that thread for a random length of time, chosen from some starting interval, in the hope that the randomness will change the relative timing of the livelocked transactions enough that on the next try one will succeed and then the other can then proceed without interference. If the transaction again encounters interference, it tries again, but on each retry not only does the lock manager choose a new random delay, but it also increases the interval from which the delay is chosen by some multiplicative constant, typically 2. Since on each retry there is an increased probability of success, one can push this probability as close to unity as desired by continued retries, with the expectation that the interfering transactions will

很难选择一个合适的计时器值，既能保持进程顺利推进，又能适应正常的延迟和操作时间的变化。如果环境或系统负载发生变化，可能需要重新调整所有这些计时器的值，这在大型系统中可能是一项相当繁琐的工作。

4. *Cycle detection*（乐观策略）。在锁管理器中维护一个等待图（如5.2.5节所述），该图显示哪些事务已获取哪些锁，以及哪些事务正在等待哪些锁。每当另一个事务尝试获取锁时，发现锁已被占用并提议等待，锁管理器会检查该图以判断等待是否会导致循环，从而引发死锁。如果会，锁管理器会选择循环中的某个成员作为牺牲品，单方面中止该事务，以便其他事务可以继续执行。被中止的事务随后会重试，希望其他事务已取得足够进展而不再阻塞，从而避免再次发生死锁。

当系统采用锁排序、回退或循环检测机制时，通常还会设置定时器作为安全网。这是因为硬件故障或编程错误（如无限循环）可能导致进度阻塞的情况，而这些情况是任何死锁检测方法都无法捕捉到的。

由于死锁检测算法可能引入额外的事务中止原因，可以设想一种极端情况：无论调用者重试多少次，该算法都会中止执行某特定事务的每一次尝试。例如，假设名为Alphonse和Gaston的两个线程在尝试获取名为Apple和Banana的两个对象的锁时陷入死锁：Alphonse先获取Apple的锁，Gaston获取Banana的锁；接着Alphonse尝试获取Banana的锁并进入等待，而Gaston尝试获取Apple的锁也陷入等待，从而形成死锁。最终，Alphonse因超时开始回滚更新以准备释放锁，与此同时Gaston也因超时执行相同操作。两者重启后再次陷入死锁，且计时器设置与之前完全相同，很可能无限重复这一过程。因此，我们仍无法确保系统能取得进展。这正是第5章所称的*live-lock*涌现特性——从形式上看从未发生实际死锁，两个线程表面上都在执行看似有用的操作。

处理活锁的一种方法是应用第七章[在线]中熟悉技术的随机化版本：*exponential random backoff*。当计时器到期导致事务中止时，锁管理器在清除锁后，会随机延迟该线程一段时间，这个时间从某个初始区间中选取，目的是通过随机性改变活锁事务的相对时序，使得下一次尝试时其中一个事务能够成功，另一个事务随后也能无干扰地继续执行。如果事务再次遭遇干扰，它会重试，但每次重试时，锁管理器不仅会选取一个新的随机延迟，还会以某个乘法常数（通常为2）扩大延迟选取的区间。由于每次重试成功的概率都会增加，通过持续重试，可以将成功概率推至任意接近1的水平，预期干扰事务最终会...

eventually get out of one another's way. A useful property of exponential random backoff is that if repeated retries continue to fail it is almost certainly an indication of some deeper problem—perhaps a programming mistake or a level of competition for shared variables that is intrinsically so high that the system should be redesigned.

The design of more elaborate algorithms or programming disciplines that guarantee progress is a project that has only modest potential payoff, and an *end-to-end argument* suggests that it may not be worth the effort. In practice, systems that would have frequent interference among transactions are not usually designed with a high degree of concurrency anyway. When interference is not frequent, simple techniques such as safety-net timers and exponential random backoff not only work well, but they usually must be provided anyway, to cope with any races or programming errors such as endless loops that may have crept into the system design or implementation. Thus a more complex progress-guaranteeing discipline is likely to be redundant, and only rarely will it get a chance to promote progress.

9.6 Atomicity across Layers and Multiple Sites

There remain some important gaps in our exploration of atomicity. First, in a layered system, a transaction implemented in one layer may consist of a series of component actions of a lower layer that are themselves atomic. The question is how the commitment of the lower-layer transactions should relate to the commitment of the higher layer transaction. If the higher-layer transaction decides to abort, the question is what to do about lower-layer transactions that may have already committed. There are two possibilities:

- Reverse the effect of any committed lower-layer transactions with an UNDO action. This technique requires that the results of the lower-layer transactions be visible only within the higher-layer transaction.
- Somehow delay commitment of the lower-layer transactions and arrange that they actually commit at the same time that the higher-layer transaction commits.

Up to this point, we have assumed the first possibility. In this section we explore the second one.

Another gap is that, as described so far, our techniques to provide atomicity all involve the use of shared variables in memory or storage (for example, pointers to the latest version, outcome records, logs, and locks) and thus implicitly assume that the composite actions that make up a transaction all occur in close physical proximity. When the composing actions are physically separated, communication delay, communication reliability, and independent failure make atomicity both more important and harder to achieve.

We will edge up on both of these problems by first identifying a common subproblem: implementing nested transactions. We will then extend the solution to the nested transaction problem to create an agreement protocol, known as *two-phase commit,* that

最终会彼此让路。指数随机退避的一个有用特性是，如果重复重试持续失败，几乎可以肯定这表明存在更深层次的问题——可能是编程错误，或者对共享变量的竞争程度本质上过高，以至于系统需要重新设计。

设计更为精细的算法或编程规范以确保进展，这一项目的潜在回报相当有限，且*end-to-end argument* 表明其可能不值得投入精力。实际上，那些交易间频繁发生干扰的系统通常本就不会设计成高度并发的模式。当干扰并不频繁时，诸如安全网计时器和指数随机退避等简单技术不仅效果良好，而且通常必须予以配备，以应对可能潜入系统设计或实现中的任何竞态条件或编程错误（如无限循环）。因此，一套更为复杂的进展保障机制很可能是冗余的，且鲜有机会真正推动进展。

9.6 跨层与多站点的原子性

在我们对原子性的探索中，仍存在一些重要的空白。首先，在分层系统中，某一层实现的事务可能由一系列较低层的组件动作构成，这些动作本身也是原子的。问题在于，较低层事务的提交应如何与较高层事务的提交相关联。如果较高层事务决定中止，那么对于那些可能已经提交的较低层事务该如何处理。这里有两种可能性：

- 通过一个UNDO 操作来逆转任何已提交的下层事务的影响。这一技术要求下层事务的结果仅在高层事务内部可见。
- 设法延迟下层事务的提交，并安排它们实际上在上层事务提交的同时完成提交。

至此，我们一直假设第一种可能性。本节我们将探讨第二种情况。

另一个差距在于，如前所述，我们目前提供的原子性技术都涉及使用内存或存储中的共享变量（例如指向最新版本的指针、结果记录、日志和锁），因此隐含地假设构成事务的复合操作都在物理上紧密相邻的位置发生。当这些组合操作在物理上分散时，通信延迟、通信可靠性以及独立故障不仅使原子性变得更加重要，也使其更难以实现。

我们将从这两个问题入手，首先识别一个共同的子问题：实现嵌套事务。接着，我们会将嵌套事务问题的解决方案扩展，以创建一种被称为*two-phase commit,*的共识协议，该协议

```
procedure PAY_INTEREST (reference account)
    if account.balance > 0 then
        interest = account.balance * 0.05
        TRANSFER (bank, account, interest)
    else
        interest = account.balance * 0.15
        TRANSFER (account, bank, interest)

procedure MONTH_END_INTEREST:()
    for A ← each customer_account do
        PAY_INTEREST (A)
```

FIGURE 9.35

An example of two procedures, one of which calls the other, yet each should be individually atomic.

coordinates commitment of lower-layer transactions. We can then extend the two-phase commit protocol, using a specialized form of remote procedure call, to coordinate steps that must be carried out at different places. This sequence is another example of bootstrapping; the special case that we know how to handle is the single-site transaction and the more general problem is the multiple-site transaction. As an additional observation, we will discover that multiple-site transactions are quite similar to, but not quite the same as, the *dilemma of the two generals*.

9.6.1 Hierarchical Composition of Transactions

We got into the discussion of transactions by considering that complex interpreters are engineered in layers, and that each layer should implement atomic actions for its next-higher, client layer. Thus transactions are nested, each one typically consisting of multiple lower-layer transactions. This nesting requires that some additional thought be given to the mechanism of achieving atomicity.

Consider again a banking example. Suppose that the TRANSFER procedure of Section 9.1.5 is available for moving funds from one account to another, and it has been implemented as a transaction. Suppose now that we wish to create the two application procedures of Figure 9.35. The first procedure, PAY_INTEREST, invokes TRANSFER to move an appropriate amount of money from or to an internal account named *bank*, the direction and rate depending on whether the customer account balance is positive or negative. The second procedure, MONTH_END_INTEREST, fulfills the bank's intention to pay (or extract) interest every month on every customer account by iterating through the accounts and invoking PAY_INTEREST on each one.

It would probably be inappropriate to have two invocations of MONTH_END_INTEREST running at the same time, but it is likely that at the same time that MONTH_END_INTEREST is running there are other banking activities in progress that are also invoking TRANSFER.

```
procedure PAY_INTEREST (reference account)
    if account.balance > 0 then
        interest = account.balance * 0.05
        TRANSFER (bank, account, interest)
    else
        interest = account.balance * 0.15
        TRANSFER (account, bank, interest)

procedure MONTH_END_INTEREST:()
    for A ← each customer_account do
        PAY_INTEREST (A)
```

图9.35

一个关于两个过程的例子，其中一个调用了另一个，但每个过程本身都应是原子性的。

协调下层事务的提交。我们可以扩展两阶段提交协议，利用一种特殊形式的远程过程调用，来协调必须在不同地点执行的步骤。这一序列是引导过程的又一例证；我们已知如何处理的特例是单点事务，而更普遍的问题则是多点事务。进一步观察可以发现，多点事务与{v*}非常相似，但并不完全相同。

9.6.1 交易的分层组合

我们之所以讨论事务，是因为考虑到复杂解释器是以分层方式设计的，每一层都应为其更高层的客户层实现原子操作。因此，事务是嵌套的，每个事务通常由多个下层事务构成。这种嵌套要求我们对实现原子性的机制给予额外的思考。

再次考虑一个银行示例。假设第9.1.5节中的TRANSFER程序可用于将资金从一个账户转移到另一个账户，并且已将其实现为事务。现在假设我们希望创建图9.35中的两个应用程序过程。第一个过程PAY_INTEREST调用TRANSFER，将适当金额的资金转入或转出一个名为*bank*的内部账户，方向和利率取决于客户账户余额是正还是负。第二个过程MONTH_END_INTEREST通过遍历所有账户并对每个账户调用PAY_INTEREST，实现银行每月对每个客户账户支付（或收取）利息的意图。

同时运行两个MONTH_END_INTEREST 的调用可能不太合适，但在MONTH_END_INTEREST 运行的同时，很可能还有其他银行业务活动正在进行，这些活动也在调用TRANSFER。

It is also possible that the **for each** statement inside MONTH_END_INTEREST actually runs several instances of its iteration (and thus of PAY_INTEREST) concurrently. Thus we have a need for three layers of transactions. The lowest layer is the TRANSFER procedure, in which debiting of one account and crediting of a second account must be atomic. At the next higher layer, the procedure PAY_INTEREST should be executed atomically, to ensure that some concurrent TRANSFER transaction doesn't change the balance of the account between the positive/negative test and the calculation of the interest amount. Finally, the procedure MONTH_END_INTEREST should be a transaction, to ensure that some concurrent TRANSFER transaction does not move money from an account A to an account B between the interest-payment processing of those two accounts, since such a transfer could cause the bank to pay interest twice on the same funds. Structurally, an invocation of the TRANSFER procedure is nested inside PAY_INTEREST, and one or more concurrent invocations of PAY_INTEREST are nested inside MONTH_END_INTEREST.

The reason nesting is a potential problem comes from a consideration of the commit steps of the nested transactions. For example, the commit point of the TRANSFER transaction would seem to have to occur either before or after the commit point of the PAY_INTEREST transaction, depending on where in the programming of PAY_INTEREST we place its commit point. Yet either of these positions will cause trouble. If the TRANSFER commit occurs in the pre-commit phase of PAY_INTEREST then if there is a system crash PAY_INTEREST will not be able to back out as though it hadn't tried to operate because the values of the two accounts that TRANSFER changed may have already been used by concurrent transactions to make payment decisions. But if the TRANSFER commit does not occur until the post-commit phase of PAY_INTEREST, there is a risk that the transfer itself can not be completed, for example because one of the accounts is inaccessible. The conclusion is that somehow the commit point of the nested transaction should coincide with the commit point of the enclosing transaction. A slightly different coordination problem applies to MONTH_END_INTEREST: no TRANSFERs by other transactions should occur while it runs (that is, it should run either before or after any concurrent TRANSFER transactions), but it must be able to do multiple TRANSFERs itself, each time it invokes PAY_INTEREST, and its own possibly concurrent transfer actions must be before-or-after actions, since they all involve the account named "bank".

Suppose for the moment that the system provides transactions with version histories. We can deal with nesting problems by extending the idea of an outcome record: we allow outcome records to be organized hierarchically. Whenever we create a nested transaction, we record in its outcome record both the initial state (PENDING) of the new transaction and the identifier of the enclosing transaction. The resulting hierarchical arrangement of outcome records then exactly reflects the nesting of the transactions. A top-layer outcome record would contain a flag to indicate that it is not nested inside any other transaction. When an outcome record contains the identifier of a higher-layer transaction, we refer to it as a *dependent* outcome record, and the record to which it refers is called its *superior*.

The transactions, whether nested or enclosing, then go about their business, and depending on their success mark their own outcome records COMMITTED or ABORTED, as usual. However, when READ_CURRENT_VALUE (described in Section 9.4.2) examines the sta-

也可能出现**for each** 语句在MONTH_END_INTEREST内部实际运行多次迭代实例（从而并发执行PAY_INTEREST)）的情况。因此，我们需要三层事务结构。最底层是TRANSFER过程，其中从一个账户扣款与向另一个账户存款必须作为原子操作执行。在更高一层，PAY_INTEREST过程应原子化执行，以确保某些并发的TRANSFER事务不会在正/负测试与利息金额计算之间更改账户余额。最终，MONTH_END_INTEREST过程应作为事务处理，以保证某些并发的TRANSFER事务不会在这两个账户的利息支付处理期间将资金从账户A转移至账户B，因为此类转账可能导致银行对同一笔资金重复支付利息。从结构上看，TRANS-FER过程的调用嵌套于PAY_INTEREST内部，而一个或多个PAY_INTEREST的并发调用则嵌套在MONTH_END_INTEREST之中。

嵌套之所以可能成为问题，源于对嵌套事务提交步骤的考量。例如，TRANSFER事务的提交点似乎必须发生在PAY_INTEREST事务提交点之前或之后，具体取决于在PAY_INTEREST的程序设计中将其提交点置于何处。然而，这两种位置都会引发问题。若TRANSFER 提交发生在PAY_INTEREST的预提交阶段，一旦系统崩溃，PAY_INTEREST将无法回滚至未尝试操作的状态，因为TRANSFER修改的两个账户值可能已被并发事务用于支付决策。但若TRANSFER提交延迟至PAY_INTEREST的后提交阶段，则存在转账本身无法完成的风险，例如因某个账户无法访问。由此得出的结论是：嵌套事务的提交点必须以某种方式与外围事务的提交点重合。 对于MONTH_END_INTEREST，则存在一个稍有不同的协调问题：在其运行期间，其他事务不得进行任何TRANSFER操作（即它应在所有并发TRANSFER事务之前或之后运行），但它自身必须能执行多次TRANSFER操作——每次调用PAY_INTEREST时皆可执行，且其自身可能并发的转账动作必须为"前或后"动作，因为它们都涉及名为“bank”的账户。

暂且假设该系统为事务提供了版本历史。我们可以通过扩展结果记录的概念来处理嵌套问题：允许结果记录按层次结构组织。每当创建一个嵌套事务时，我们会在其结果记录中同时记录新事务的初始状态（PENDING）以及外层事务的标识符。这样形成的结果记录层次结构便能精确反映事务的嵌套关系。顶层结果记录会包含一个标志，表明其不嵌套于任何其他事务内。若某结果记录包含更高层事务的标识符，我们称其为*dependent*结果记录，而其所引用的记录则称为其*superior*。

事务，无论是嵌套的还是外层的，都会照常进行它们的操作，并根据成功与否标记各自的结果记录COMMITTED或ABORTED。然而，当READ_CURRENT_VALUE（如第9.4.2节所述）检查状态时——

tus of a version to see whether or not the transaction that created it is COMMITTED, it must additionally check to see if the outcome record contains a reference to a superior outcome record. If so, it must follow the reference and check the status of the superior. If that record says that it, too, is COMMITTED, it must continue following the chain upward, if necessary all the way to the highest-layer outcome record. The transaction in question is actually COMMITTED only if all the records in the chain are in the COMMITTED state. If any record in the chain is ABORTED, this transaction is actually ABORTED, despite the COMMITTED claim in its own outcome record. Finally, if neither of those situations holds, then there must be one or more records in the chain that are still PENDING. The outcome of this transaction remains PENDING until those records become COMMITTED or ABORTED. Thus the outcome of an apparently-COMMITTED dependent outcome record actually depends on the outcomes of all of its ancestors. We can describe this situation by saying that, until all its ancestors commit, this lower-layer transaction is sitting on a knife-edge, at the point of committing but still capable of aborting if necessary. For purposes of discussion we will identify this situation as a distinct virtual state of the outcome record and the transaction, by saying that the transaction is *tentatively committed*.

This hierarchical arrangement has several interesting programming consequences. If a nested transaction has any post-commit steps, those steps cannot proceed until all of the hierarchically higher transactions have committed. For example, if one of the nested transactions opens a cash drawer when it commits, the sending of the release message to the cash drawer must somehow be held up until the highest-layer transaction determines its outcome.

This output visibility consequence is only one example of many relating to the tentatively committed state. The nested transaction, having declared itself tentatively committed, has renounced the ability to abort—the decision is in someone else's hands. It must be able to run to completion *or* to abort, and it must be able to maintain the tentatively committed state indefinitely. Maintaining the ability to go either way can be awkward, since the transaction may be holding locks, keeping pages in memory or tapes mounted, or reliably holding on to output messages. One consequence is that a designer cannot simply take any arbitrary transaction and blindly use it as a nested component of a larger transaction. At the least, the designer must review what is required for the nested transaction to maintain the tentatively committed state.

Another, more complex, consequence arises when one considers possible interactions among different transactions that are nested within the same higher-layer transaction. Consider our earlier example of TRANSFER transactions that are nested inside PAY_INTEREST, which in turn is nested inside MONTH_END_INTEREST. Suppose that the first time that MONTH_END_INTEREST invokes PAY_INTEREST, that invocation commits, thus moving into the tentatively committed state, pending the outcome of MONTH_END_INTEREST. Then MONTH_END_INTEREST invokes PAY_INTEREST on a second bank account. PAY_INTEREST needs to be able to read as input data the value of the bank's own interest account, which is a pending result of the previous, tentatively COMMITTED, invocation of PAY_INTEREST. The READ_CURRENT_VALUE algorithm, as implemented in Section 9.4.2, doesn't distinguish between reads arising within the same group of nested transactions and reads from some

要查看某个版本的交易状态是否为COMMITTED，必须额外检查其成果记录是否引用了上级成果记录。若存在引用，则需追溯该引用并检查上级记录的状态。若该记录同样显示为COMMITTED，则需继续沿链向上追溯，必要时直至最高层成果记录。只有当链中所有记录均处于COMMITTED状态时，该交易才真正为COMMITTED。若链中任一记录为ABORTED，则无论其自身成果记录如何声明COMMITTED，该交易实际为ABORTED。若上述情况均不适用，则链中必存在一个或多个仍为PENDING的记录。此时该交易的成果将保持PENDING状态，直至这些记录转为COMMITTED或ABORTED。因此，表面看似COMMITTED的依赖型成果记录，其实际结果取决于所有祖先记录的结果。我们可以这样描述：在所有祖先提交前，该低层交易如同立于刀锋之上——处于即将提交的状态，但仍可能在必要时中止。为便于讨论，我们将此情形定义为成果记录和交易的一个独特虚拟状态，称之为 *tentatively committed*。

这种层次化安排带来了几个有趣的编程影响。如果嵌套事务包含任何提交后步骤，这些步骤必须等到所有更高层次的事务都提交后才能执行。例如，若某个嵌套事务在提交时触发了钱箱开启操作，那么向钱箱发送解锁指令的动作必须被暂缓，直至最外层事务确定其最终结果。

这一输出可见性后果仅是众多与暂态提交状态相关的例子之一。嵌套事务在声明自身为暂态提交后，便放弃了中止的能力——决定权已掌握在他人手中。它必须能够运行至完成*or*才能中止，同时必须能够无限期地维持暂态提交状态。保持这种进退两可的状态可能颇为棘手，因为事务可能正持有锁、将页面保留在内存中或磁带挂载状态，或是可靠地持有输出消息。一个必然结果是，设计者不能简单地选取任意事务，盲目地将其用作更大事务的嵌套组件。至少，设计者必须审查嵌套事务维持暂态提交状态所需的条件。

另一个更为复杂的后果出现在考虑同一高层事务内嵌套的不同事务之间可能的交互时。以我们之前的例子为例，TRANSFER事务嵌套在PAY_INTEREST内，而PAY_INTEREST又嵌套在MONTH_END_INTEREST中。假设MONTH_END_INTEREST首次调用PAY_INTEREST时，该调用成功提交，从而进入暂定提交状态，等待MONTH_END_INTEREST的结果。接着，MONTH_END_INTEREST在第二个银行账户上调用PAY_INTEREST。PAY_INTEREST需要能够读取银行自身利息账户的值作为输入数据，这是之前PAY_INTEREST的暂定COMMITTED,调用的待定结果。如第9.4.2节所实现的READ_CURRENT_VALUE算法，并未区分来自同一组嵌套事务内部的读取与来自其他事务的读取。

completely unrelated transaction. Figure 9.36 illustrates the situation. If the test in READ_CURRENT_VALUE for committed values is extended by simply following the ancestry of the outcome record controlling the latest version, it will undoubtedly force the second invocation of PAY_INTEREST to wait pending the final outcome of the first invocation of PAY_INTEREST. But since the outcome of that first invocation depends on the outcome of

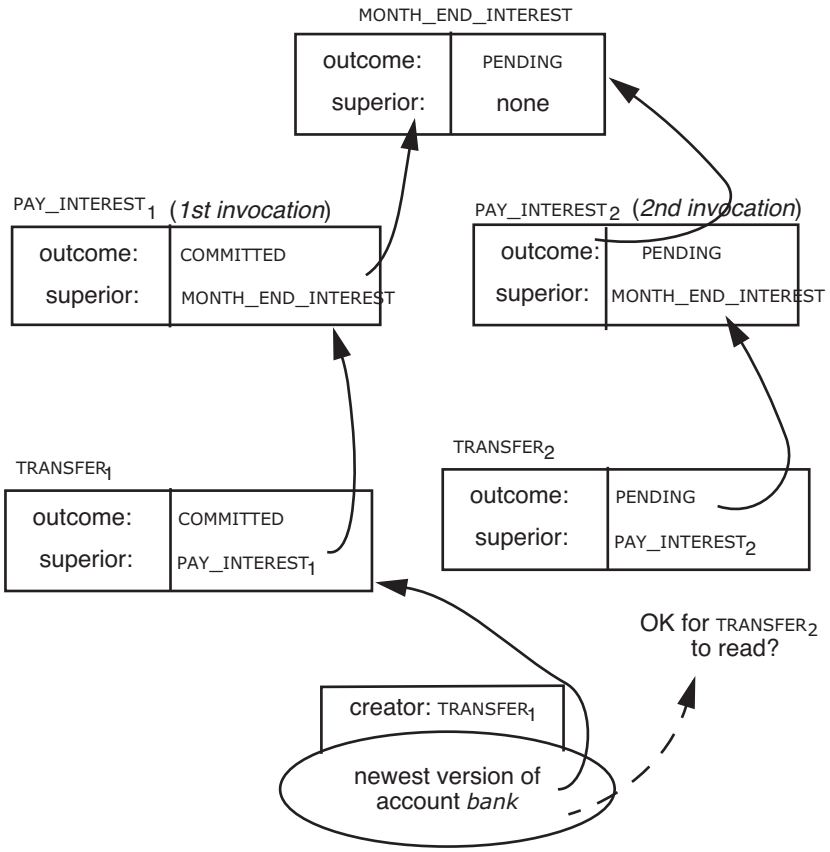

FIGURE 9.36

Transaction TRANSFER_2, nested in transaction PAY_INTEREST_2, which is nested in transaction MONTH_END_INTEREST, wants to read the current value of account *bank*. But *bank* was last written by transaction TRANSFER_1, which is nested in COMMITTED transaction PAY_INTEREST_1, which is nested in still-PENDING transaction MONTH_END_INTEREST. Thus this version of *bank* is actually PENDING, rather than COMMITTED as one might conclude by looking only at the outcome of TRANSFER_1. However, TRANSFER_1 and TRANSFER_2 share a common ancestor (namely, MONTH_END_INTEREST), and the chain of transactions leading from bank to that common ancestor is completely committed, so the read of *bank* can—and to avoid a deadlock, must—be allowed.

完全不相关的交易。图9.36展示了这种情况。如果READ_CURRENT_VALUE 中对已提交值的测试通过简单地跟随控制最新版本的结果记录的祖先链来扩展，无疑会迫使PAY_INTEREST的第二次调用等待PAY_INTEREST第一次调用的最终结果。但由于第一次调用的结果又依赖于

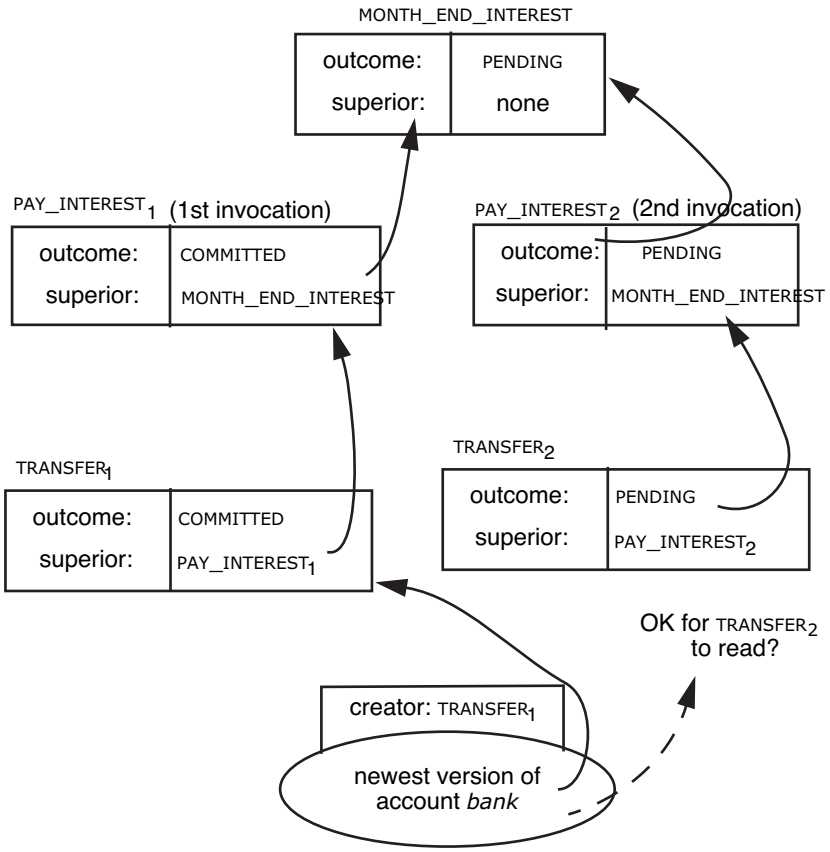

图9.36

事务TRANSFER2嵌套在事务PAY_INTEREST2中，而后者又嵌套在事务MONTH_END_INTEREST内，它想要读取账户*bank*的当前值。但*bank*最后一次是由事务TRANSFER1写入的，该事务嵌套在COMMITTED事务PAY_INTEREST1中，而后者又嵌套在仍在进行中的PENDING事务MONTH_END_INTEREST内。因此，*bank*的这个版本实际上是PENDING，而非仅通过观察TRANSFER1的结果可能得出的COMMITTED 。然而，TRANSFER1和TRANSFER2拥有共同的祖先（即MONTH_END_INTEREST），且从银行到该共同祖先的事务链已完全提交，因此对*bank* 的读取可以——并且为避免死锁，必须——被允许。

MONTH_END_INTEREST, and the outcome of MONTH_END_INTEREST currently depends on the success of the second invocation of PAY_INTEREST, we have a built-in cycle of waits that at best can only time out and abort.

Since blocking the read would be a mistake, the question of when it might be OK to permit reading of data values created by tentatively COMMITTED transactions requires some further thought. The before-or-after atomicity requirement is that no update made by a tentatively COMMITTED transaction should be visible to any transaction that would survive if for some reason the tentatively COMMITTED transaction ultimately aborts. Within that constraint, updates of tentatively COMMITTED transactions can freely be passed around. We can achieve that goal in the following way: compare the outcome record ancestry of the transaction doing the read with the ancestry of the outcome record that controls the version to be read. If these ancestries do not merge (that is, there is no common ancestor) then the reader must wait for the version's ancestry to be completely committed. If they do merge and all the transactions in the ancestry of the data version that are below the point of the merge are tentatively committed, no wait is necessary. Thus, in Figure 9.36, MONTH_END_INTEREST might be running the two (or more) invocations of PAY_INTEREST concurrently. Each invocation will call CREATE_NEW_VERSION as part of its plan to update the value of account "bank", thereby establishing a serial order of the invocations. When later invocations of PAY_INTEREST call READ_CURRENT_VALUE to read the value of account "bank", they will be forced to wait until all earlier invocations of PAY_INTEREST decide whether to commit or abort.

9.6.2 Two-Phase Commit

Since a higher-layer transaction can comprise several lower-layer transactions, we can describe the commitment of a hierarchical transaction as involving two distinct phases. In the first phase, known variously as the *preparation* or *voting* phase, the higher-layer transaction invokes some number of distinct lower-layer transactions, each of which either aborts or, by committing, becomes tentatively committed. The top-layer transaction evaluates the situation to establish that all (or enough) of the lower-layer transactions are tentatively committed that it can declare the higher-layer transaction a success.

Based on that evaluation, it either COMMITs or ABORTs the higher-layer transaction. Assuming it decides to commit, it enters the second, *commitment* phase, which in the simplest case consists of simply changing its own state from PENDING to COMMITTED or ABORTED. If it is the highest-layer transaction, at that instant all of the lower-layer tentatively committed transactions also become either COMMITTED or ABORTED. If it is itself nested in a still higher-layer transaction, it becomes tentatively committed and its component transactions continue in the tentatively committed state also. We are implementing here a coordination protocol known as *two-phase commit*. When we implement multiple-site atomicity in the next section, the distinction between the two phases will take on additional clarity.

月末利息，且当前MONTH_END_INTEREST的结果取决于第二次调用PAY_INTEREST的成功与否，我们内置了一个等待循环，最多只能超时并中止。

由于阻止读取将是一个错误，何时允许读取由暂定COMMITTED事务创建的数据值才合适，这一问题需要进一步思考。原子性的前后要求是：任何可能存活的交易都不应看到由暂定COMMITTED事务做出的更新，以防该暂定COMMITTED事务最终因故中止。在此约束下，暂定COMMITTED事务的更新可以自由传递。我们可以通过以下方式实现这一目标：比较执行读取的事务的结果记录祖先与控制待读取版本的结果记录的祖先。若这些祖先未合并（即无共同祖先），则读取者必须等待该版本的祖先完全提交。若它们确实合并，且数据版本祖先中合并点以下的所有事务均为暂定提交，则无需等待。因此，在图9.36中，MONTH_END_INTEREST可能同时运行两个（或更多）PAY_INTEREST的调用。每次调用将作为其更新“bank”账户值计划的一部分调用CREATE_NEW_VERSION，从而建立调用的串行顺序。当后续PAY_INTEREST 调用READ_CURRENT_VALUE 读取“bank”账户值时，它们将被迫等待，直到所有先前PAY_INTEREST 的调用决定提交或中止。

9.6.2 两阶段提交

由于高层事务可由多个低层事务组成，我们可以将分层事务的提交描述为包含两个不同阶段。第一阶段，即所谓的*preparation*或*voting*阶段，高层事务会调用若干独立的低层事务，每个低层事务要么中止，要么通过提交进入暂态提交状态。顶层事务通过评估当前状况，确认所有（或足够数量）低层事务已暂态提交后，即可宣告高层事务成功完成。

基于该评估，它要么COMMIT要么ABORT更高层的事务。假设它决定提交，便进入第二个*commitment*阶段，在最简单的情况下，这一阶段仅需将其自身状态从PENDING更改为COMMITTED或ABORTED。若其为最高层事务，则在此刻所有较低层暂态提交的事务也将变为COMMITTED或ABORTED。若其本身嵌套于更高层事务中，则其转为暂态提交状态，其组成事务同样保持暂态提交状态。此处我们实现了一种称为*two-phase commit*的协调协议。下一节实现多站点原子性时，两阶段之间的区别将更加清晰。

If the system uses version histories for atomicity, the hierarchy of Figure 9.36 can be directly implemented by linking outcome records. If the system uses logs, a separate table of pending transactions can contain the hierarchy, and inquiries about the state of a transaction would involve examining this table.

The concept of nesting transactions hierarchically is useful in its own right, but our particular interest in nesting is that it is the first of two building blocks for multiple-site transactions. To develop the second building block, we next explore what makes multiple-site transactions different from single-site transactions.

9.6.3 Multiple-Site Atomicity: Distributed Two-Phase Commit

If a transaction requires executing component transactions at several sites that are separated by a best-effort network, obtaining atomicity is more difficult because any of the messages used to coordinate the transactions of the various sites can be lost, delayed, or duplicated. In Chapter 4 we learned of a method, known as Remote Procedure Call (RPC) for performing an action at another site. In Chapter 7[on-line] we learned how to design protocols such as RPC with a persistent sender to ensure at-least-once execution and duplicate suppression to ensure at-most-once execution. Unfortunately, neither of these two assurances is exactly what is needed to ensure atomicity of a multiple-site transaction. However, by properly combining a two-phase commit protocol with persistent senders, duplicate suppression, and single-site transactions, we can create a correct multiple-site transaction. We assume that each site, on its own, is capable of implementing local transactions, using techniques such as version histories or logs and locks for all-or-nothing atomicity and before-or-after atomicity. Correctness of the multiple-site atomicity protocol will be achieved if all the sites commit or if all the sites abort; we will have failed if some sites commit their part of a multiple-site transaction while others abort their part of that same transaction.

Suppose the multiple-site transaction consists of a coordinator Alice requesting component transactions X, Y, and Z of worker sites Bob, Charles, and Dawn, respectively. The simple expedient of issuing three remote procedure calls certainly does not produce a transaction for Alice because Bob may do X while Charles may report that he cannot do Y. Conceptually, the coordinator would like to send three messages, to the three workers, like this one to Bob:

```
From: Alice
To: Bob
Re: my transaction 91

    if (Charles does Y and Dawn does Z) then do X, please.
```

and let the three workers handle the details. We need some clue how Bob could accomplish this strange request.

The clue comes from recognizing that the coordinator has created a higher-layer transaction and each of the workers is to perform a transaction that is nested in the higher-layer transaction. Thus, what we need is a distributed version of the two-phase commit protocol. The complication is that the coordinator and workers cannot reliably

如果系统采用版本历史记录来保证原子性，图9.36中的层级结构可以直接通过链接结果记录来实现。若系统采用日志机制，则可通过一个独立的待处理事务表来维护该层级结构，查询事务状态时需检查此表。

嵌套事务的层次化概念本身就有其价值，但我们特别关注嵌套的原因在于，它是构建多站点事务的两大基石中的第一块。为了构建第二块基石，接下来我们将探讨多站点事务与单站点事务的本质区别。

9.6.3 多站点原子性：分布式两阶段提交

如果一项事务需要在多个由尽力而为网络分隔的站点上执行组件事务，那么实现原子性将更为困难，因为用于协调各站点事务的任何消息都可能丢失、延迟或重复。在第四章中，我们了解了一种名为远程过程调用（RPC）的方法，用于在另一站点执行操作。第七章[在线]中，我们学习了如何设计如RPC这样的协议，配合持久化发送方以确保至少一次执行，并通过重复抑制确保至多一次执行。遗憾的是，这两种保证都无法精确满足确保多站点事务原子性的需求。然而，通过恰当地结合两阶段提交协议与持久化发送方、重复抑制及单站点事务，我们能够构建出正确的多站点事务。我们假设每个站点自身能够利用版本历史、日志、锁等技术实现本地事务，确保全有或全无的原子性及前后一致的原子性。若所有站点均提交或均中止，则多站点原子性协议的正确性即得以实现；反之，若部分站点提交其多站点事务部分而其他站点中止同一事务的对应部分，则意味着协议执行失败。

假设多站点事务由协调者Alice发起，她分别向工作站点Bob、Charles和Dawn请求执行组件事务X、Y和Z。简单地发出三个远程过程调用显然无法为Alice构成一个完整的事务，因为Bob可能执行了X，而Charles却报告无法完成Y。从概念上讲，协调者需要向三位工作者发送三条消息，例如发给Bob的消息如下：

> 来自：Alice　致：Bob　关于：我的交易91
>
> 如果（Charles做Y且Dawn做Z）那么请做X。

让那三位工人去处理细节。我们需要一些线索来了解鲍勃是如何完成这个奇怪请求的。

线索在于认识到协调者创建了一个更高层的事务，而每个工作者需要执行一个嵌套在该高层事务中的事务。因此，我们需要的是一种分布式版本的两阶段提交协议。复杂之处在于协调者与工作者无法可靠地

communicate. The problem thus reduces to constructing a reliable distributed version of the two-phase commit protocol. We can do that by applying persistent senders and duplicate suppression.

Phase one of the protocol starts with coordinator Alice creating a top-layer outcome record for the overall transaction. Then Alice begins persistently sending to Bob an RPC-like message:

```
From:Alice
To: Bob
Re: my transaction 271

    Please do X as part of my transaction.
```

Similar messages go from Alice to Charles and Dawn, also referring to transaction 271, and requesting that they do Y and Z, respectively. As with an ordinary remote procedure call, if Alice doesn't receive a response from one or more of the workers in a reasonable time she resends the message to the non-responding workers as many times as necessary to elicit a response.

A worker site, upon receiving a request of this form, checks for duplicates and then creates a transaction of its own, but it makes the transaction a *nested* one, with its superior being Alice's original transaction. It then goes about doing the pre-commit part of the requested action, reporting back to Alice that this much has gone well:

```
From:Bob
To: Alice
Re: your transaction 271

    My part X is ready to commit.
```

Alice, upon collecting a complete set of such responses then moves to the two-phase commit part of the transaction, by sending messages to each of Bob, Charles, and Dawn saying, e.g.:

```
Two-phase-commit message #1:

    From:Alice
    To: Bob
    Re: my transaction 271

        PREPARE to commit X.
```

Bob, upon receiving this message, commits—but only tentatively—or aborts. Having created durable tentative versions (or logged to journal storage its planned updates) and having recorded an outcome record saying that it is PREPARED either to commit or abort, Bob then persistently sends a response to Alice reporting his state:

因此，问题就转化为构建一个可靠的两阶段提交协议的分布式版本。我们可以通过应用持久化发送者和重复抑制机制来实现这一点。

协议的第一阶段始于协调者Alice为整个交易创建一个顶层结果记录。然后，Alice开始持续向Bob发送类似RPC的消息：

发件人：Alice　收件人Bob　主题：关于我的交易271　请作为我交易的一部分执行X。

类似的消息从爱丽丝发送给查尔斯和道恩，同样提及交易271，并分别要求他们执行Y和Z操作。如同普通的远程过程调用一样，如果爱丽丝在合理时间内未收到一个或多个工作者的响应，她会向未响应的工作者重新发送消息，必要时重复多次，直至获得回应。

一个工作站点在收到此类请求时，会先检查是否有重复，然后创建自己的事务，但将其设为*nested*类型，其上级事务为爱丽丝的原始事务。接着，该站点会执行请求操作的预提交部分，并向爱丽丝反馈此部分进展顺利：

发件人：Bob　收件人Alice
主题：关于您的交易271
我的X部分已准备好提交。

爱丽丝在收集到一整套这样的响应后，便进入事务的两阶段提交环节，向鲍勃、查尔斯和道恩分别发送消息，例如：

T两阶段提交消息 #1:
来自：Alice　致Bob　关于：我的交易271　准备提交。

鲍勃在收到这条消息后，会进行提交——但只是试探性的——或者中止。在创建了持久的试探性版本（或将计划更新记录到日志存储中）并记录了一个结果条目，表明其状态为PREPARED（无论是提交还是中止）之后，鲍勃会持久地向爱丽丝发送一个响应，报告他的状态：

Two-phase-commit message #2:

From:Bob
To:Alice
Re: your transaction 271

I am PREPARED to commit my part. Have you decided to commit yet? Regards.

or alternatively, a message reporting it has aborted. If Bob receives a duplicate request from Alice, his persistent sender sends back a duplicate of the PREPARED or ABORTED response.

At this point Bob, being in the PREPARED state, is out on a limb. Just as in a local hierarchical nesting, Bob must be able either to run to the end or to abort, to maintain that state of preparation indefinitely, and wait for someone else (Alice) to say which. In addition, the coordinator may independently crash or lose communication contact, increasing Bob's uncertainty. If the coordinator goes down, all of the workers must wait until it recovers; in this protocol, the coordinator is a single point of failure.

As coordinator, Alice collects the response messages from her several workers (perhaps re-requesting PREPARED responses several times from some worker sites). If all workers send PREPARED messages, phase one of the two-phase commit is complete. If any worker responds with an abort message, or doesn't respond at all, Alice has the usual choice of aborting the entire transaction or perhaps trying a different worker site to carry out that component transaction. Phase two begins when Alice commits the entire transaction by marking her own outcome record COMMITTED.

Once the higher-layer outcome record is marked as COMMITTED or ABORTED, Alice sends a completion message back to each of Bob, Charles, and Dawn:

Two-phase-commit message #3

From:Alice
To:Bob
Re: my transaction 271

My transaction committed. Thanks for your help.

Each worker site, upon receiving such a message, changes its state from PREPARED to COMMITTED, performs any needed post-commit actions, and exits. Meanwhile, Alice can go about other business, with one important requirement for the future: she must remember, reliably and for an indefinite time, the outcome of this transaction. The reason is that one or more of her completion messages may have been lost. Any worker sites that are in the PREPARED state are awaiting the completion message to tell them which way to go. If a completion message does not arrive in a reasonable period of time, the persistent sender at the worker site will resend its PREPARED message. Whenever Alice receives a duplicate PREPARED message, she simply sends back the current state of the outcome record for the named transaction.

If a worker site that uses logs and locks crashes, the recovery procedure at that site has to take three extra steps. First, it must classify any PREPARED transaction as a tentative winner that it should restore to the PREPARED state. Second, if the worker is using locks for

两阶段提交消息 #2： 发件人：Bob 收件人：Alice 主题：关于您的交易271 我已准备就绪（PREPARED）提交我方的部分。您是否已决定提交？此致敬礼。

或者，也可以是一条报告已中止的消息。如果Bob从Alice那里收到重复的请求，他的持久发送方会返回一个PREPARED或ABORTED 响应的副本。

此时，处于PREPARED状态的Bob陷入了孤立无援的境地。正如在局部层次嵌套中那样，Bob必须能够选择执行到底或中止操作，以无限期维持这种准备状态，并等待他人（如Alice）做出决定。此外，协调者可能独立崩溃或失去通信联系，这进一步加剧了Bob的不确定性。若协调者宕机，所有工作节点都必须等待其恢复；在该协议中，协调者构成了单点故障。

作为协调者，Alice从她的多个工作者那里收集响应消息（可能还会多次向某些工作站点重新请求PREPARED响应）。如果所有工作者都发送了PREPARED消息，两阶段提交的第一阶段就完成了。如果有任何工作者回复了中止消息，或者根本没有响应，Alice通常可以选择中止整个事务，或者尝试让另一个工作站点来执行该组件事务。第二阶段开始于Alice通过标记她自己的结果记录COMMITTED来提交整个事务。

一旦高层结果记录被标记为COMMITTED或ABORTED，Alice就会向Bob、Charles和Dawn各自发送一条完成消息：

两阶段提交消息 #3

来自：Alice 致Bob 主题：关于我的交易271 我的交易已提交。感谢你的帮助。

每个工作站点在接收到此类消息后，会将其状态从PREPARED更改为COM-MITTED，执行任何必要的提交后操作，然后退出。与此同时，Alice可以继续处理其他事务，但未来有一个重要要求：她必须可靠且无限期地记住此次事务的结果。原因在于，她发送的一个或多个完成消息可能已经丢失。处于PREPARED状态的任何工作站点都在等待完成消息以确定后续操作方向。若完成消息在合理时间内未到达，工作站点上的持久发送方将重新发送其PREPARED消息。每当Alice收到重复的PREPARED消息时，她只需返回该命名事务结果记录的当前状态即可。

如果一个使用日志和锁的工作站点崩溃，该站点的恢复过程需要采取三个额外步骤。首先，它必须将任何PREPARED事务归类为应恢复至PREPARED状态的暂定胜者。其次，如果该工作者正在使用锁来

before-or-after atomicity, the recovery procedure must reacquire any locks the PREPARED transaction was holding at the time of the failure. Finally, the recovery procedure must restart the persistent sender, to learn the current status of the higher-layer transaction. If the worker site uses version histories, only the last step, restarting the persistent sender, is required.

Since the workers act as persistent senders of their PREPARED messages, Alice can be confident that every worker will eventually learn that her transaction committed. But since the persistent senders of the workers are independent, Alice has no way of ensuring that they will act simultaneously. Instead, Alice can only be certain of eventual completion of her transaction. This distinction between simultaneous action and eventual action is critically important, as will soon be seen.

If all goes well, two-phase commit of *N* worker sites will be accomplished in *3N* messages, as shown in Figure 9.37: for each worker site a PREPARE message, a PREPARED message in response, and a COMMIT message. This *3N* message protocol is complete and sufficient, although there are several variations one can propose.

An example of a simplifying variation is that the initial RPC request and response could also carry the PREPARE and PREPARED messages, respectively. However, once a worker sends a PREPARED message, it loses the ability to unilaterally abort, and it must remain on the knife edge awaiting instructions from the coordinator. To minimize this wait, it is usually preferable to delay the PREPARE/PREPARED message pair until the coordinator knows that the other workers seem to be in a position to do their parts.

Some versions of the distributed two-phase commit protocol have a fourth acknowledgment message from the worker sites to the coordinator. The intent is to collect a complete set of acknowledgment messages—the coordinator persistently sends completion messages until every site acknowledges. Once all acknowledgments are in, the coordinator can then safely discard its outcome record, since every worker site is known to have gotten the word.

A system that is concerned both about outcome record storage space and the cost of extra messages can use a further refinement, called *presumed commit*. Since one would expect that most transactions commit, we can use a slightly odd but very space-efficient representation for the value COMMITTED of an outcome record: non-existence. The coordinator answers any inquiry about a non-existent outcome record by sending a COMMITTED response. If the coordinator uses this representation, it commits by destroying the outcome record, so a fourth acknowledgment message from every worker is unnecessary. In return for this apparent magic reduction in both message count and space, we notice that outcome records for aborted transactions can not easily be discarded because if an inquiry arrives after discarding, the inquiry will receive the response COMMITTED. The coordinator can, however, persistently ask for acknowledgment of aborted transactions, and discard the outcome record after all these acknowledgments are in. This protocol that leads to discarding an outcome record is identical to the protocol described in Chapter 7[on-line] to close a stream and discard the record of that stream.

Distributed two-phase commit does not solve all multiple-site atomicity problems. For example, if the coordinator site (in this case, Alice) is aboard a ship that sinks after

在前后原子性操作中，恢复过程必须重新获取PREPARED 事务在故障发生时持有的所有锁。最后，恢复过程必须重启持久化发送器，以了解高层事务的当前状态。如果工作站点使用了版本历史记录，则仅需执行最后一步——重启持久化发送器。

由于工作者们作为其PREPARED消息的持久发送方持续运作，Alice可以确信每个工作者终将获悉她的交易已提交。然而，由于这些工作者的持久发送机制彼此独立，Alice无法确保它们会同步行动。她只能确定交易最终会完成。这种同步行动与最终行动之间的区别至关重要，这一点即将得到阐明。

如果一切顺利，*N*工作节点的两阶段提交将通过*3N*条消息完成，如图9.37所示：每个工作节点发送一条PREPARE消息，接收一条PREPARED响应消息，再发送一条COMMIT消息。尽管可以提出若干变体方案，但这一*3N*消息协议已是完备且充分的。

简化变体的一个例子是，初始的RPC请求和响应也可以分别携带PREPARE和PREPARED消息。然而，一旦工作节点发送了PREPARED 消息，它就失去了单方面中止的能力，必须保持在临界状态，等待协调器的指令。为了最小化这种等待，通常更可取的做法是延迟PREPARE/PREPARED消息对，直到协调器确认其他工作节点似乎已准备好执行各自的任务。

分布式两阶段提交协议的某些版本包含从工作站点到协调者的第四条确认消息。其目的是收集完整的确认消息集——协调者会持续发送完成消息，直至每个站点都予以确认。一旦所有确认到位，协调者便可安全地丢弃其结果记录，因为已知每个工作站点都已收到通知。

一个既关注结果记录存储空间又在意额外消息成本的系统，可以采用名为*presumed commit*的进一步优化方案。考虑到大多数事务预期会提交，我们可以为结果记录的值COMMITTED 采用一种略显奇特但极其节省空间的表示方式——不存在性。协调器通过发送COMMITTED 响应来答复任何关于不存在结果记录的查询。若协调器采用此表示法，它通过销毁结果记录来完成提交，因此无需来自每个工作者的第四条确认消息。作为这种在消息数量和空间上看似神奇减少的代价，我们注意到已中止事务的结果记录不易被丢弃，因为若在丢弃后收到查询，该查询将收到COMMITTED响应。然而，协调器可以持续要求对中止事务进行确认，并在收到所有确认后丢弃结果记录。这种导致结果记录被丢弃的协议，与第7章[在线]描述的关闭流并丢弃该流记录的协议完全相同。

分布式两阶段提交并不能解决所有多站点的原子性问题。例如，如果协调站点（此处为Alice）位于一艘沉没的船上，在{v*}之后

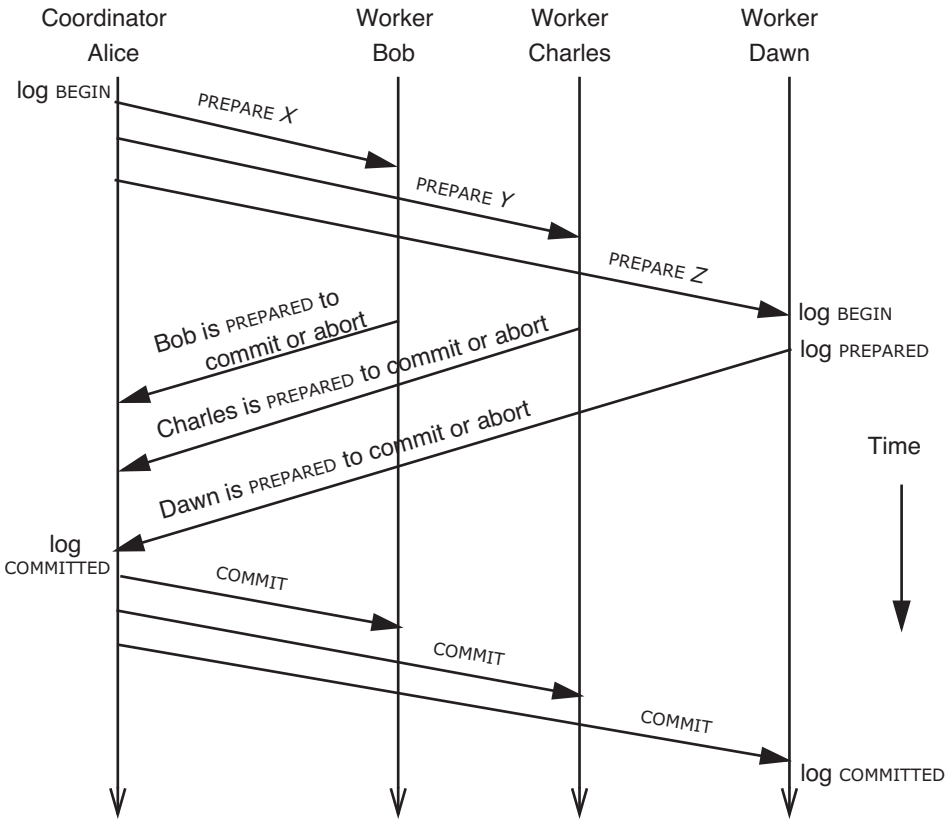

FIGURE 9.37

Timing diagram for distributed two-phase commit, using *3N* messages. (The initial RPC request and response messages are not shown.) Each of the four participants maintains its own version history or recovery log. The diagram shows log entries made by the coordinator and by one of the workers.

sending the PREPARE message but before sending the COMMIT or ABORT message the worker sites are in left in the PREPARED state with no way to proceed. Even without that concern, Alice and her co-workers are standing uncomfortably close to a multiple-site atomicity problem that, at least in principle, can *not* be solved. The only thing that rescues them is our observation that the several workers will do their parts eventually, not necessarily simultaneously. If she had required simultaneous action, Alice would have been in trouble.

The unsolvable problem is known as the *dilemma of the two generals*.

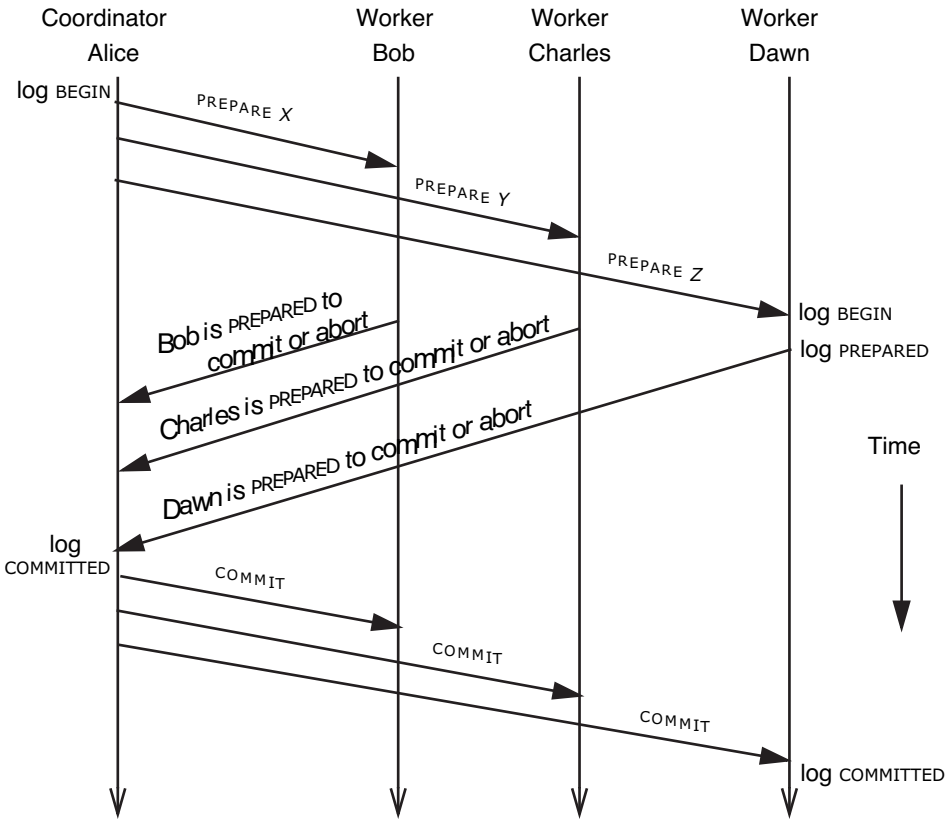

图9.37

分布式两阶段提交的时序图，使用3N条消息。（初始的RPC请求和响应消息未显示。）四位参与者各自维护其版本历史或恢复日志。图中展示了协调者与其中一位工作者的日志条目记录。

在发送PREPARE消息之后，但在发送COMMIT或ABORT消息之前，工作站点被遗留在了PREPARED状态，无法继续前进。即便不考虑这一点，Alice和她的同事们也正尴尬地站在一个多站点原子性问题边缘，这个问题至少在理论上*not*是可以解决的。唯一解救他们的是我们观察到，多个工作者最终会完成各自的部分，但不一定同时进行。如果Alice要求同时行动，那她可就麻烦大了。

这个无法解决的问题被称为*dilemma of the two generals.*

9.6.4 The Dilemma of the Two Generals

An important constraint on possible coordination protocols when communication is unreliable is captured in a vivid analogy, called the *dilemma of the two generals*.* Suppose that two small armies are encamped on two mountains outside a city. The city is well-enough defended that it can repulse and destroy either one of the two armies. Only if the two armies attack simultaneously can they take the city. Thus the two generals who command the armies desire to coordinate their attack.

The only method of communication between the two generals is to send runners from one camp to the other. But the defenders of the city have sentries posted in the valley separating the two mountains, so there is a chance that the runner, trying to cross the valley, will instead fall into enemy hands, and be unable to deliver the message.

Suppose that the first general sends this message:

```
From:Julius Caesar
To:Titus Labienus
Date:11 January

I propose to cross the Rubicon and attack at dawn tomorrow. OK?
```

expecting that the second general will respond either with:

```
From:Titus Labienus
To:Julius Caesar;
Date:11 January

Yes, dawn on the 12th.
```

or, possibly:

```
From:Titus Labienus
To:Julius Caesar
Date:11 January

No. I am awaiting reinforcements from Gaul.
```

Suppose further that the first message does not make it through. In that case, the second general does not march because no request to do so arrives. In addition, the first general does not march because no response returns, and all is well (except for the lost runner).

Now, instead suppose the runner delivers the first message successfully and second general sends the reply "Yes," but that the reply is lost. The first general cannot distinguish this case from the earlier case, so that army will not march. The second general has agreed to march, but knowing that the first general won't march unless the "Yes" confirmation arrives, the second general will not march without being certain that the first

* The origin of this analogy has been lost, but it was apparently first described in print in 1977 by Jim N. Gray in his "Notes on Database Operating Systems", reprinted in *Operating Systems, Lecture Notes in Computer Science 60*, Springer Verlag, 1978. At about the same time, Danny Cohen described another analogy he called the dating protocol, which is congruent with the dilemma of the two generals.

9.6.4 两将军难题

当通信不可靠时，对可能的协调协议的一个重要约束可以用一个生动的类比来概括，称为*dilemma of the two generals*。*假设两支小部队驻扎在城外两座山上。该城防御坚固，足以击退并摧毁其中任何一支军队。只有当两支军队同时发起进攻时，他们才能攻占该城。因此，指挥这两支军队的两位将军希望协调他们的进攻行动。

两位将军之间唯一的通讯方式，便是派遣信使往返于两座营地之间。然而，城中的守军在山谷间布下了哨兵，这片山谷将两座山脉分隔开来。因此，信使在试图穿越山谷时，有可能落入敌手，导致信息无法送达。

假设第一位将军发送了这条消息：

> 来自：尤利乌斯·凯撒
> 致：提图斯·拉比努斯
> 日期：1月11日
>
> 我提议渡过卢比孔河，明日拂晓发起进攻。可否?

期待第二位将军会以以下方式回应：

> 发自：提图斯·拉比努斯　致：尤利乌斯·凯撒；　日期：1月11日　是的，12日的黎明。

或者，可能：

> 发自：提图斯·拉比努斯　致：尤利乌斯·凯撒　日期1月11日　不。我正在等待来自高卢的增援。

进一步假设第一条消息未能成功传达。在这种情况下，由于未收到任何行动请求，第二位将军不会进军。此外，由于没有回应返回，第一位将军也不会采取行动，一切相安无事（除了那名不幸的信使）。

现在，假设跑者成功传递了第一条消息，而第二位将军发出了“同意”的回复，但该回复丢失了。第一位将军无法将这种情况与之前的情况区分开来，因此军队不会行动。第二位将军已同意行动，但知道除非“同意”的确认送达，否则第一位将军不会行动，所以第二位将军在没有确定第一位将军会行动的情况下也不会贸然行动。

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* The origin of this analogy has been lost, but it was apparently first described in print in 1977 by Jim N. Gray in his “Notes on Database Operating Systems”, reprinted in *Operating Systems, Lecture Notes in Computer Science 60*, Springer Verlag, 1978. At about the same time, Danny Cohen described another analogy he called the dating protocol, which is congruent with the dilemma of the two generals.

general received the confirmation. This hesitation on the part of the second general suggests that the first general should send back an acknowledgment of receipt of the confirmation:

> From:Julius Caesar
> To:Titus Labienus
> Date:11 January
>
> The die is cast.

Unfortunately, that doesn't help, since the runner carrying this acknowledgment may be lost and the second general, not receiving the acknowledgment, will still not march. Thus the dilemma.

We can now leap directly to a conclusion: there is no protocol with a bounded number of messages that can convince both generals that it is safe to march. If there were such a protocol, the *last* message in any particular run of that protocol must be unnecessary to safe coordination because it might be lost, undetectably. Since the last message must be unnecessary, one could delete that message to produce another, shorter sequence of messages that must guarantee safe coordination. We can reapply the same reasoning repeatedly to the shorter message sequence to produce still shorter ones, and we conclude that if such a safe protocol exists it either generates message sequences of zero length or else of unbounded length. A zero-length protocol can't communicate anything, and an unbounded protocol is of no use to the generals, who must choose a particular time to march.

A practical general, presented with this dilemma by a mathematician in the field, would reassign the mathematician to a new job as a runner, and send a scout to check out the valley and report the probability that a successful transit can be accomplished within a specified time. Knowing that probability, the general would then send several (hopefully independent) runners, each carrying a copy of the message, choosing a number of runners large enough that the probability is negligible that all of them fail to deliver the message before the appointed time. (The loss of all the runners would be what Chapter 8[on-line] called an intolerable error.) Similarly, the second general sends many runners each carrying a copy of either the "Yes" or the "No" acknowledgment. This procedure provides a practical solution of the problem, so the dilemma is of no real consequence. Nevertheless, it is interesting to discover a problem that cannot, in principle, be solved with complete certainty.

We can state the theoretical conclusion more generally and succinctly: if messages may be lost, no bounded protocol can guarantee with complete certainty that both generals know that they will both march at the same time. The best that they can do is accept some non-zero probability of failure equal to the probability of non-delivery of their last message.

It is interesting to analyze just why we can't we use a distributed two-phase commit protocol to resolve the dilemma of the two generals. As suggested at the outset, it has to do with a subtle difference in *when* things may, or must, happen. The two generals require, in order to vanquish the defenses of the city, that they march at the *same* time.

将军收到了确认。第二位将军的犹豫表明，第一位将军应当发回一份确认收到的回执：

> 来自：尤利乌斯·凯撒
> 致：提图斯·拉比努斯
> 日期1月11日　骰子已掷出。

遗憾的是，这无济于事，因为携带这一确认信息的信使可能会丢失，而第二位将军未收到确认，依然不会进军。这就是困境所在。

我们现在可以直接得出一个结论：不存在一种消息数量有限的协议，能够同时让两位将军确信发起进攻是安全的。如果存在这样的协议，那么在该协议的任何一次具体执行中，*last*消息对于安全协调而言必然是非必需的，因为它可能无法察觉地丢失。既然最后一条消息必须是非必需的，那么我们可以删除这条消息，从而产生另一个更短的消息序列，该序列仍必须保证安全协调。我们可以对缩短后的消息序列重复应用相同的推理，进一步缩短序列长度，最终得出结论：若存在这样一种安全协议，它要么生成零长度的消息序列，要么生成无限长度的序列。零长度的协议无法传递任何信息，而无限长度的协议对将军们毫无用处，因为他们必须选择一个具体的进攻时机。

一位务实的将军，在战场上遇到数学家提出的这一困境时，会将该数学家调任为传令兵的新职务，并派遣侦察兵去探查山谷，报告在指定时间内成功穿越的概率。得知这一概率后，将军将派出多名（最好是互不影响的）传令兵，每人携带一份信息副本，所选传令兵的数量要足够多，以确保所有人在约定时间前都未能送达信息的概率微乎其微。（所有传令兵全部失联的情况，将被视为第8章[在线]所述的不可容忍错误。）同理，第二位将军也会派出多名传令兵，每人携带“是”或“否”的确认回复副本。这一流程为问题提供了实际解决方案，因此该困境并无实质影响。然而，发现一个在原则上无法完全确定解决的问题，仍不失为一件趣事。

我们可以更普遍且简洁地阐述这一理论结论：如果消息可能丢失，那么任何有限的协议都无法完全确保两位将军能同时确认他们将一起进军。他们所能达到的最佳效果，是接受一个非零的失败概率，这个概率等于他们最后一条消息未能送达的概率。

分析为何不能采用分布式两阶段提交协议来解决两位将军的困境，这一探讨颇为有趣。正如最初所指出的，关键在于*when*事情可能发生或必须发生的微妙差异。两位将军为了攻克城池的防御，必须在*same*同时进军。

The persistent senders of the distributed two-phase commit protocol ensure that if the coordinator decides to commit, all of the workers will eventually also commit, but there is no assurance that they will do so at the same time. If one of the communication links goes down for a day, when it comes back up the worker at the other end of that link will then receive the notice to commit, but this action may occur a day later than the actions of its colleagues. Thus the problem solved by distributed two-phase commit is slightly relaxed when compared with the dilemma of the two generals. That relaxation doesn't help the two generals, but the relaxation turns out to be just enough to allow us to devise a protocol that ensures correctness.

By a similar line of reasoning, there is no way to ensure with complete certainty that actions will be taken simultaneously at two sites that communicate only via a best-effort network. Distributed two-phase commit can thus safely open a cash drawer of an ATM in Tokyo, with confidence that a computer in Munich will eventually update the balance of that account. But if, for some reason, it is necessary to open two cash drawers at different sites at the same time, the only solution is either the probabilistic approach or to somehow replace the best-effort network with a reliable one. The requirement for reliable communication is why real estate transactions and weddings (both of which are examples of two-phase commit protocols) usually occur with all of the parties in one room.

9.7 A More Complete Model of Disk Failure (Advanced Topic)

Section 9.2 of this chapter developed a failure analysis model for a calendar management program in which a system crash may corrupt at most one disk sector—the one, if any, that was being written at the instant of the crash. That section also developed a masking strategy for that problem, creating all-or-nothing disk storage. To keep that development simple, the strategy ignored decay events. This section revisits that model, considering how to also mask decay events. The result will be all-or-nothing durable storage, meaning that it is both all-or-nothing in the event of a system crash and durable in the face of decay events.

9.7.1 Storage that is Both All-or-Nothing and Durable

In Chapter 8[on-line] we learned that to obtain durable storage we should write two or more replicas of each disk sector. In the current chapter we learned that to recover from a system crash while writing a disk sector we should never overwrite the previous version of that sector, we should write a new version in a different place. To obtain storage that is both durable and all-or-nothing we combine these two observations: make more than one replica, and don't overwrite the previous version. One easy way to do that would be to simply build the all-or-nothing storage layer of the current chapter on top of the durable storage layer of Chapter 8[on-line]. That method would certainly work but it is a bit heavy-handed: with a replication count of just two, it would lead to allo-

分布式两阶段提交协议中的持久发送方确保，如果协调者决定提交，所有工作节点最终也将提交，但并不保证它们会同时进行。如果其中一条通信链路中断一天，当它恢复时，位于该链路另一端的工作节点随后会收到提交通知，但这一动作可能比其同事们的动作晚一天发生。因此，分布式两阶段提交所解决的问题与两将军困境相比略有放宽。这种放宽对两将军并无帮助，但事实证明，这种放宽恰好足以让我们设计出一个确保正确性的协议。

通过类似的推理方式，我们无法完全确保仅通过尽力而为网络通信的两个站点能同步执行操作。因此，分布式两阶段提交可以安全地打开东京一台ATM的现金抽屉，并确信慕尼黑的计算机终将更新该账户余额。但若因某些原因必须同时在不同地点开启两个现金抽屉，唯一解决方案要么采用概率性方法，要么设法用可靠网络替代尽力而为网络。房地产交易和婚礼（两者均为两阶段提交协议的实例）通常要求所有相关方共处一室，正是出于对可靠通信的需求。

9.7 更完整的磁盘故障模型（高级主题）

本章第9.2节为日历管理程序开发了一个故障分析模型，该系统崩溃时最多可能损坏一个磁盘扇区——即崩溃瞬间正在被写入的那个扇区（如果有的话）。该节还针对该问题提出了一种掩蔽策略，实现了全有或全无的磁盘存储。为了使该方案保持简洁，策略中未考虑衰变事件。本节将重新审视该模型，探讨如何同时掩蔽衰变事件。最终将实现全有或全无的持久存储，这意味着它既能在系统崩溃时保持全有或全无特性，又能在衰变事件面前保持持久性。

9.7.1 全有或全无且持久的存储

在第8章[在线]中，我们了解到要实现持久存储，应当为每个磁盘扇区写入两个或多个副本。而在本章中，我们认识到要从写入磁盘扇区时的系统崩溃中恢复，绝不应覆盖该扇区的旧版本，而应在不同位置写入新版本。为了获得既持久又具备全有或全无特性的存储，我们将这两点观察结合起来：创建多个副本，并且不覆盖先前版本。一个简单的方法就是直接在第八章[在线]的持久存储层之上构建本章的全有或全无存储层。这种方法固然可行，但略显粗放：仅以副本数为二为例，它会导致存——

cating six disk sectors for each sector of real data. This is a case in which modularity has an excessive cost.

Recall that the parameter that Chapter 8[on-line] used to determine frequency of checking the integrity of disk storage was the expected time to decay, T_d. Suppose for the moment that the durability requirement can be achieved by maintaining only two copies. In that case, T_d must be much greater than the time required to write two copies of a sector on two disks. Put another way, a large T_d means that the short-term chance of a decay event is small enough that the designer may be able to safely neglect it. We can take advantage of this observation to devise a slightly risky but far more economical method of implementing storage that is both durable and all-or-nothing with just two replicas. The basic idea is that if we are confident that we have two good replicas of some piece of data for durability, it is safe (for all-or-nothing atomicity) to overwrite one of the two replicas; the second replica can be used as a backup to ensure all-or-nothing atomicity if the system should happen to crash while writing the first one. Once we are confident that the first replica has been correctly written with new data, we can safely overwrite the second one, to regain long-term durability. If the time to complete the two writes is short compared with T_d, the probability that a decay event interferes with this algorithm will be negligible. Figure 9.38 shows the algorithm and the two replicas of the data, here named *D0* and *D1*.

An interesting point is that ALL_OR_NOTHING_DURABLE_GET does not bother to check the status returned upon reading *D1*—it just passes the status value along to its caller. The reason is that in the absence of decay CAREFUL_GET has *no* expected errors when reading data that CAREFUL_PUT was allowed to finish writing. Thus the returned status would be BAD only in two cases:

1. CAREFUL_PUT of *D1* was interrupted in mid-operation, or
2. *D1* was subject to an unexpected decay.

The algorithm guarantees that the first case cannot happen. ALL_OR_NOTHING_DURABLE_PUT doesn't begin CAREFUL_PUT on data *D1* until after the completion of its CAREFUL_PUT on data *D0*. At most one of the two copies could be BAD because of a system crash during CAREFUL_PUT. Thus if the first copy (*D0*) is BAD, then we expect that the second one (*D1*) is OK.

The risk of the second case is real, but we have assumed its probability to be small: it arises only if there is a random decay of *D1* in a time much shorter than T_d. In reading *D1* we have an opportunity to *detect* that error through the status value, but we have no way to recover when both data copies are damaged, so this detectable error must be classified as untolerated. All we can do is pass a status report along to the application so that it knows that there was an untolerated error.

There is one currently unnecessary step hidden in the SALVAGE program: if *D0* is BAD, nothing is gained by copying *D1* onto *D0*, since ALL_OR_NOTHING_DURABLE_PUT, which called SALVAGE, will immediately overwrite *D0* with new data. The step is included because it allows SALVAGE to be used in a refinement of the algorithm.

为每个实际数据扇区分配六个磁盘扇区。这是模块化带来过高成本的一个例子。

回想一下，第八章[在线]用于确定磁盘存储完整性检查频率的参数是预期衰减时间T_d。假设目前仅需维护两份副本即可满足耐久性要求，那么T_d必须远大于在两块磁盘上写入一个扇区的两份副本所需时间。换言之，较大的T_d意味着短期内发生衰减事件的概率足够低，设计者可以安全地忽略它。我们可以利用这一观察结果，设计一种略带风险但经济得多的存储实现方法——仅用两个副本就能同时满足持久性和全有或全无特性。其核心理念是：若确信某数据有两个完好副本保障持久性，那么覆盖其中一个是安全的（就全有或全无原子性而言）；若系统在写入第一个副本时意外崩溃，第二个副本可作为备份确保全有或全无原子性。一旦确认第一个副本已正确写入新数据，便可安全覆盖第二个副本以恢复长期耐久性。若两次写入完成时间相较于T_d足够短，衰减事件干扰该算法的概率将微乎其微。图9.38展示了该算法及数据的两个副本，此处命名为*D0*和*D1*。

一个有趣的点是，ALL_OR_NOTHING_DURABLE_GET并不费心去检查读取*D1*时返回的状态——它只是将状态值传递给其调用者。原因在于，在没有衰减的情况下，CAREFUL_GET在读取CAREFUL_PUT已完成写入的数据时，预期会有*no*个错误。因此，返回的状态BAD仅在两种情况下出现：

1. CAREFUL_PUT的*D1*在操作过程中被中断，或
2. *D1* 遭遇了意外的衰减。

该算法确保第一种情况不会发生。ALL_OR_NOTHING_DURABLE_PUT在数据*D1*上不会开始CAREFUL_PUT，直到其在数据*D0*上的CAREFUL_PUT完成。由于在CAREFUL_PUT期间系统崩溃，最多只有其中一个副本可能被BAD。因此，如果第一个副本（*D0*）被BAD，那么我们预期第二个副本（*D1*）是OK。

第二种情况的风险确实存在，但我们假设其发生的概率很小：只有当*D1* 在远短于T_d的时间内发生随机衰减时才会出现。在读取*D1*时，我们有机会通过状态值*detect*来纠正该错误，但当两份数据副本都受损时，我们便无计可施，因此这种可检测到的错误必须被归类为不可容忍错误。我们所能做的就是向应用程序传递一份状态报告，让其知晓发生了不可容忍错误。

在SALVAGE程序中隐藏着一个目前不必要的步骤：如果*D0*等于BAD，将*D1*复制到*D0*上并无实际收益，因为调用SALVAGE的ALL_OR_NOTHING_DURABLE_PUT会立即用新数据覆盖*D0*。保留这一步骤的原因是它允许SALVAGE在算法的优化版本中被使用。

In the absence of decay events, this algorithm would be just as good as the all-or-nothing procedures of Figures 9.6 and 9.7, and it would perform somewhat better, since it involves only two copies. Assuming that errors are rare enough that recovery operations do not dominate performance, the usual cost of ALL_OR_NOTHING_DURABLE_GET is just one disk read, compared with three in the ALL_OR_NOTHING_GET algorithm. The cost of ALL_OR_NOTHING_DURABLE_PUT is two disk reads (in SALVAGE) and two disk writes, compared with three disk reads and three disk writes for the ALL_OR_NOTHING_PUT algorithm.

That analysis is based on a decay-free system. To deal with decay events, thus making the scheme both all-or-nothing *and* durable, the designer adopts two ideas from the discussion of durability in Chapter 8[on-line], the second of which eats up some of the better performance:

1. Place the two copies, *D0* and *D1*, in independent decay sets (for example write them on two different disk drives, preferably from different vendors).
2. Have a clerk run the SALVAGE program on every atomic sector at least once every T_d seconds.

```
1  procedure ALL_OR_NOTHING_DURABLE_GET (reference data, atomic_sector)
2      ds ← CAREFUL_GET (data, atomic_sector.D0)
3      if ds = BAD then
4          ds ← CAREFUL_GET (data, atomic_sector.D1)
5      return ds

6  procedure ALL_OR_NOTHING_DURABLE_PUT (new_data, atomic_sector)
7      SALVAGE(atomic_sector)
8      ds ← CAREFUL_PUT (new_data, atomic_sector.D0)
9      ds ← CAREFUL_PUT (new_data, atomic_sector.D1)
10     return ds

11 procedure SALVAGE(atomic_sector)        //Run this program every T_d seconds.
12     ds0 ← CAREFUL_GET (data0, atomic_sector.D0)
13     ds1 ← CAREFUL_GET (data1, atomic_sector.D1)
14     if ds0 = BAD then
15         CAREFUL_PUT (data1, atomic_sector.D0)
16     else if ds1 = BAD then
17         CAREFUL_PUT (data0, atomic_sector.D1)
18     if data0 ≠ data1 then
19         CAREFUL_PUT (data0, atomic_sector.D1)
```

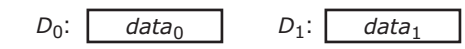


FIGURE 9.38

Data arrangement and algorithms to implement all-or-nothing durable storage on top of the careful storage layer of Figure 8.12.

在没有衰减事件的情况下，该算法将与图9.6和9.7中的全有或全无程序同样有效，且由于仅涉及两份副本，其表现会略胜一筹。假设错误发生频率足够低，以至于恢复操作不会主导性能，ALL_OR_NOTHING_DURABLE_GET的常规成本仅为一次磁盘读取，而ALL_OR_NOTHING_GET算法则需要三次。ALL_OR_NOTHING_DURABLE_PUT的成本为两次磁盘读取（在SALVAGE中）和两次磁盘写入，相比之下，ALL_OR_NOTHING_PUT算法需要三次磁盘读取和三次磁盘写入。

该分析基于一个无衰减系统。为了处理衰减事件，从而使方案具备全有或全无*and*的持久性，设计者采用了第8章[在线]关于持久性讨论中的两个观点，其中第二个观点会消耗部分更优的性能：

1. 将两份副本*D0*和*D1*分别存放在独立的衰变集中（例如将它们写入两个不同的磁盘驱动器，最好来自不同厂商）。
2. 让一名职员至少每T_d秒在每个原子扇区上运行一次SALVAGE程序。

```
1   procedure ALL_OR_NOTHING_DURABLE_GET (reference data, atomic_sector)
2       ds ← CAREFUL_GET (data, atomic_sector.D0)
3       if ds = BAD then
4           ds ← CAREFUL_GET (data, atomic_sector.D1)
5       return ds

6   procedure ALL_OR_NOTHING_DURABLE_PUT (new_data, atomic_sector)
7       SALVAGE(atomic_sector)
8       ds ← CAREFUL_PUT (new_data, atomic_sector.D0)
9       ds ← CAREFUL_PUT (new_data, atomic_sector.D1)
10      return ds

11  procedure SALVAGE(atomic_sector)        //Run this program every T_d seconds.
12      ds0 ← CAREFUL_GET (data0, atomic_sector.D0)
13      ds1 ← CAREFUL_GET (data1, atomic_sector.D1)
14      if ds0 = BAD then
15          CAREFUL_PUT (data1, atomic_sector.D0)
16      else if ds1 = BAD then
17          CAREFUL_PUT (data0, atomic_sector.D1)
18      if data0 ≠ data1 then
19          CAREFUL_PUT (data0, atomic_sector.D1)
```

图9.38

数据排列与算法，用于在图8.12的谨慎存储层之上实现全有或全无的持久存储。

The clerk running the SALVAGE program performs $2N$ disk reads every T_d seconds to maintain N durable sectors. This extra expense is the price of durability against disk decay. The performance cost of the clerk depends on the choice of T_d, the value of N, and the priority of the clerk. Since the expected operational lifetime of a hard disk is usually several years, setting T_d to a few weeks should make the chance of untolerated failure from decay negligible, especially if there is also an operating practice to routinely replace disks well before they reach their expected operational lifetime. A modern hard disk with a capacity of one terabyte would have about $N = 10^9$ kilobyte-sized sectors. If it takes 10 milliseconds to read a sector, it would take about 2×10^7 seconds, or two days, for a clerk to read all of the contents of two one-terabyte hard disks. If the work of the clerk is scheduled to occur at night, or uses a priority system that runs the clerk when the system is otherwise not being used heavily, that reading can spread out over a few weeks and the performance impact can be minor.

A few paragraphs back mentioned that there is the potential for a refinement: If we also run the SALVAGE program on every atomic sector immediately following every system crash, then it should not be necessary to do it at the beginning of every ALL_OR_NOTHING_DURABLE_PUT. That variation, which is more economical if crashes are infrequent and disks are not too large, is due to Butler Lampson and Howard Sturgis [Suggestions for Further Reading 1.8.7]. It raises one minor concern: it depends on the rarity of coincidence of two failures: the spontaneous decay of one data replica at about the same time that CAREFUL_PUT crashes in the middle of rewriting the other replica of that same sector. If we are convinced that such a coincidence is rare, we can declare it to be an untolerated error, and we have a self-consistent and more economical algorithm. With this scheme the cost of ALL_OR_NOTHING_DURABLE_PUT reduces to just two disk writes.

9.8 Case Studies: Machine Language Atomicity

9.8.1 Complex Instruction Sets: The General Electric 600 Line

In the early days of mainframe computers, most manufacturers reveled in providing elaborate instruction sets, without paying much attention to questions of atomicity. The General Electric 600 line, which later evolved to be the Honeywell Information System, Inc., 68 series computer architecture, had a feature called "indirect and tally." One could specify this feature by setting to ON a one-bit flag (the "tally" flag) stored in an unused high-order bit of any indirect address. The instruction

```
Load register A from Y indirect.
```

was interpreted to mean that the low-order bits of the cell with address Y contain another address, called an indirect address, and that indirect address should be used to retrieve the operand to be loaded into register A. In addition, if the tally flag in cell Y is ON, the processor is to increment the indirect address in Y by one and store the result back in Y. The idea is that the next time Y is used as an indirect address it will point to a different

运行SALVAGE程序的职员每T_d秒执行2N次磁盘读取，以维护N个持久化扇区。这一额外开销是为了抵御磁盘衰变而付出的持久性代价。职员的性能成本取决于T_d的选择、N的值以及职员的优先级。由于硬盘的预期使用寿命通常为几年，将T_d设置为数周应能使因衰变导致的不可容忍故障概率微乎其微，尤其是在运营实践中还会在磁盘远未达到预期寿命前就定期更换的情况下。一块容量为一太字节的现代硬盘大约包含N = 109个千字节大小的扇区。若读取一个扇区需10毫秒，则职员读取两块一太字节硬盘的全部内容约需2×10^7秒，即两天时间。若将职员的工作安排在夜间执行，或采用优先级系统在系统空闲时运行该程序，则读取操作可分散在数周内完成，对性能的影响微乎其微。

前文曾提到存在一种优化的可能性：如果在每次系统崩溃后立即对每个原子扇区运行SALVAGE程序，那么就不必在每次ALL_OR_NOTHING_DURABLE_PUT开始时进行这一操作。这种变体由巴特勒·兰普森和霍华德·斯特吉斯[进一步阅读建议1.8.7]提出，在崩溃不频繁且磁盘容量不大的情况下更为经济。但它引发了一个小问题：该方案依赖于两种故障极少同时发生——即一个数据副本自然衰变的同时，CAREFUL_PUT在重写该扇区另一副本的过程中崩溃。若能确信此类巧合极为罕见，我们可将其归为不可容忍的误差，从而获得一个自洽且更经济的算法。此方案下，ALL_OR_NOTHING_DURABLE_PUT的成本可降至仅两次磁盘写入。

9.8 案例研究：机器语言原子性

9.8.1 复杂指令集：通用电气600系列

在大型计算机的早期，大多数制造商热衷于提供复杂的指令集，却很少关注原子性问题。通用电气600系列——后来演变为霍尼韦尔信息系统公司（Honeywell Information System, Inc.）的68系列计算机架构——拥有一项名为“间接与计数”的功能。通过将任意间接地址中未使用的高位比特（即“计数”标志位）设置为ON，即可启用该功能。该指令

从Y间接加载寄存器A。

被解释为意味着地址为Y的单元的低位部分包含另一个地址，称为间接地址，该间接地址应被用来获取要加载到寄存器A中的操作数。此外，如果单元Y中的计数标志为ON，处理器会将Y中的间接地址加一，并将结果存回Y。其目的是当下次Y被用作间接地址时，它将指向一个不同的

operand—the one in the next sequential address in memory. Thus the indirect and tally feature could be used to sweep through a table. The feature seemed useful to the designers, but it was actually only occasionally, because most applications were written in higher-level languages and compiler writers found it hard to exploit. On the other hand the feature gave no end of trouble when virtual memory was retrofitted to the product line.

Suppose that virtual memory is in use, and that the indirect word is located in a page that is in primary memory, but the actual operand is in another page that has been removed to secondary memory. When the above instruction is executed, the processor will retrieve the indirect address in *Y*, increment it, and store the new value back in *Y*. Then it will attempt to retrieve the actual operand, at which time it discovers that it is not in primary memory, so it signals a missing-page exception. Since it has already modified the contents of *Y* (and by now *Y* may have been read by another processor or even removed from memory by the missing-page exception handler running on another processor), it is not feasible to back out and act as if this instruction had never executed. The designer of the exception handler would like to be able to give the processor to another thread by calling a function such as AWAIT while waiting for the missing page to arrive. Indeed, processor reassignment may be the only way to assign a processor to retrieve the missing page. However, to reassign the processor it is necessary to save its current execution state. Unfortunately, its execution state is "half-way through the instruction last addressed by the program counter." Saving this state and later restarting the processor in this state is challenging. The indirect and tally feature was just one of several sources of atomicity problems that cropped up when virtual memory was added to this processor.

The virtual memory designers desperately wanted to be able to run other threads on the interrupted processor. To solve this problem, they extended the definition of the current program state to contain not just the next-instruction counter and the program-visible registers, but also the complete internal state description of the processor—a 216-bit snapshot in the middle of the instruction. By later restoring the processor state to contain the previously saved values of the next-instruction counter, the program-visible registers, and the 216-bit internal state snapshot, the processor could exactly continue from the point at which the missing-page alert occurred. This technique worked but it had two awkward side effects: 1) when a program (or programmer) inquires about the current state of an interrupted processor, the state description includes things not in the programmer's interface; and 2) the system must be careful when restarting an interrupted program to make certain that the stored micro-state description is a valid one. If someone has altered the state description the processor could try to continue from a state it could never have gotten into by itself, which could lead to unplanned behavior, including failures of its memory protection features.

9.8.2 More Elaborate Instruction Sets: The IBM System/370

When IBM developed the System/370 by adding virtual memory to its System/360 architecture, certain System/360 multi-operand character-editing instructions caused

操作数——内存中下一个连续地址中的那个。因此，间接寻址与计数特性可用于遍历整个表格。这一特性对设计者而言似乎很有用，但实际上仅偶尔派上用场，因为大多数应用程序是用高级语言编写的，而编译器编写者发现难以利用该特性。另一方面，当产品线后期引入虚拟内存时，这一特性却带来了无尽的麻烦。

假设正在使用虚拟内存，且间接字位于主存中的一个页面内，而实际的操作数却位于已被移出至辅助存储器的另一个页面。当执行上述指令时，处理器会从*Y*中获取间接地址，对其进行递增，并将新值存回*Y*。随后，它将尝试获取实际的操作数，此时发现该操作数不在主存中，于是触发缺页异常。由于处理器已经修改了*Y*的内容（而此时*Y*可能已被其他处理器读取，甚至被运行在其他处理器上的缺页异常处理程序从内存中移除），回退并假装该指令从未执行已不可行。　　异常处理程序的设计者希望在等待缺失页面加载期间，能够通过调用诸如AWAIT这样的函数将处理器让渡给其他线程。实际上，重新分配处理器可能是唯一能将处理器指派去获取缺失页面的方式。然而，要重新分配处理器，必须保存其当前的执行状态。不幸的是，其执行状态正处于“程序计数器最后指向的指令执行到一半”的状态。保存这一状态并在之后从此状态重启处理器颇具挑战性。间接计数特性只是虚拟内存引入该处理器后涌现的多个原子性问题来源之一。

虚拟内存设计者迫切希望在中断的处理器上运行其他线程。为解决这一问题，他们扩展了当前程序状态的定义，使其不仅包含下一条指令计数器及程序可见寄存器，还囊括了处理器的完整内部状态描述——即指令执行过程中的216位快照。通过后续恢复处理器状态至先前保存的下一条指令计数器值、程序可见寄存器值及216位内部状态快照，处理器能够精确地从触发缺页警报的点继续执行。该技术虽有效，却带来两个棘手副作用：1）当程序（或程序员）查询中断处理器的当前状态时，状态描述会包含超出程序员接口范畴的内容；2）系统在重启中断程序时必须确保存储的微状态描述是有效的。若有人篡改状态描述，处理器可能试图从一个自身永远无法进入的状态继续执行，这将导致非预期行为，包括内存保护功能的失效。

9.8.2 更复杂的指令集：IBM System/370

当IBM通过在其System/360架构上添加虚拟内存来开发System/370时，某些System/360的多操作数字符编辑指令导致了

atomicity problems. For example, the TRANSLATE instruction contains three arguments, two of which are addresses in memory (call them *string* and *table*) and the third of which, *length*, is an 8-bit count that the instruction interprets as the length of *string*. TRANSLATE takes one byte at a time from *string*, uses that byte as an offset in *table*, retrieves the byte at the offset, and replaces the byte in *string* with the byte it found in *table*. The designers had in mind that TRANSLATE could be used to convert a character string from one character set to another.

The problem with adding virtual memory is that both *string* and *table* may be as long as 65,536 bytes, so either or both of those operands may cross not just one, but several page boundaries. Suppose just the first page of *string* is in physical memory. The TRANSLATE instruction works its way through the bytes at the beginning of string. When it comes to the end of that first page, it encounters a missing-page exception. At this point, the instruction cannot run to completion because data it requires is missing. It also cannot back out and act as if it never started because it has modified data in memory by overwriting it. After the virtual memory manager retrieves the missing page, the problem is how to restart the half-completed instruction. If it restarts from the beginning, it will try to convert the already-converted characters, which would be a mistake. For correct operation, the instruction needs to continue from where it left off.

Rather than tampering with the program state definition, the IBM processor designers chose a *dry run* strategy in which the TRANSLATE instruction is executed using a hidden copy of the program-visible registers and making no changes in memory. If one of the operands causes a missing-page exception, the processor can act as if it never tried the instruction, since there is no program-visible evidence that it did. The stored program state shows only that the TRANSLATE instruction is about to be executed. After the processor retrieves the missing page, it restarts the interrupted thread by trying the TRANSLATE instruction from the beginning again, another dry run. If there are several missing pages, several dry runs may occur, each getting one more page into primary memory. When a dry run finally succeeds in completing, the processor runs the instruction once more, this time for real, using the program-visible registers and allowing memory to be updated. Since the System/370 (at the time this modification was made) was a single-processor architecture, there was no possibility that another processor might snatch a page away after the dry run but before the real execution of the instruction. This solution had the side effect of making life more difficult for a later designer with the task of adding multiple processors.

9.8.3 The Apollo Desktop Computer and the Motorola M68000 Microprocessor

When Apollo Computer designed a desktop computer using the Motorola 68000 microprocessor, the designers, who wanted to add a virtual memory feature, discovered that the microprocessor instruction set interface was not atomic. Worse, because it was constructed entirely on a single chip it could not be modified to do a dry run (as in the IBM 370) or to make it store the internal microprogram state (as in the General Electric 600 line). So the Apollo designers used a different strategy: they installed not one, but two

原子性问题。例如，TRANSLATE指令包含三个参数，其中两个是内存中的地址（称为*string*和*table*），第三个参数*length*是一个8位计数，该指令将其解释为*string*的长度。TRANSLATE 每次从*string*中取出一个字节，将该字节作为*table*中的偏移量，获取偏移量处的字节，并用其在*table*中找到的字节替换*string* 中的字节。设计者设想TRANSLATE可用于将字符串从一种字符集转换为另一种字符集。

添加虚拟内存的问题在于，*string*和*table*都可能长达65,536字节，因此这两个操作数中的任意一个或两者都可能跨越不止一个，而是多个页面边界。假设只有*string*的第一个页面在物理内存中。TRANS-LATE 指令会逐个处理字符串开头的字节。当它到达该第一页的末尾时，会遇到缺页异常。此时，由于所需数据缺失，指令无法完成执行。它也无法回退并表现得好像从未开始执行，因为它已经通过覆写修改了内存中的数据。在虚拟内存管理器获取缺失的页面后，问题在于如何重新启动这个半完成的指令。如果从头开始重启，它将尝试转换已经转换过的字符，这将是一个错误。为了正确操作，指令需要从中断处继续执行。

IBM处理器的设计者并未选择篡改程序状态定义，而是采取了一种*dry run*策略：利用程序可见寄存器的隐藏副本执行TRANSLATE 指令，且不对内存做任何更改。若某个操作数引发缺页异常，处理器可表现得仿佛从未尝试执行该指令，因为程序可见层面没有任何执行痕迹。存储的程序状态仅显示TRANSLATE指令即将被执行。当处理器获取缺失页后，它会通过从头重试TRANSLATE 指令来恢复被中断的线程——这又是一次空运行。若存在多个缺失页，可能会发生多次空运行，每次都将一个新增页调入主存。当某次空运行最终成功完成时，处理器会再次执行该指令，这次才是真实执行：使用程序可见寄存器并允许更新内存。由于System/370（进行此修改时）是单处理器架构，不存在其他处理器在空运行后、指令真实执行前抢走页面的可能性。该解决方案的副作用是给后来承担多处理器扩展任务的设计者增加了难度。

9.8.3 阿波罗桌面计算机与摩托罗拉M68000微处理器

当阿波罗计算机公司采用摩托罗拉68000微处理器设计一款桌面电脑时，希望加入虚拟内存功能的设计师们发现，该微处理器的指令集接口并非原子操作。更糟的是，由于该处理器完全集成在单一芯片上，无法像IBM 370那样进行试运行，也无法像通用电气600系列那样存储内部微程序状态。因此，阿波罗的设计师们采取了不同的策略：他们安装的不是一个，而是两个

Motorola 68000 processors. When the first one encounters a missing-page exception, it simply stops in its tracks, and waits for the operand to appear. The second Motorola 68000 (whose program is carefully planned to reside entirely in primary memory) fetches the missing page and then restarts the first processor.

Other designers working with the Motorola 68000 used a different, somewhat risky trick: modify all compilers and assemblers to generate only instructions that happen to be atomic. Motorola later produced a version of the 68000 in which all internal state registers of the microprocessor could be saved, the same method used in adding virtual memory to the General Electric 600 line.

Exercises

9.1 *Locking up humanities*: The registrar's office is upgrading its scheduling program for limited-enrollment humanities subjects. The plan is to make it multithreaded, but there is concern that having multiple threads trying to update the database at the same time could cause trouble. The program originally had just two operations:

```
status ← REGISTER (subject_name)
DROP (subject_name)
```

where *subject_name* was a string such as "21W471". The REGISTER procedure checked to see if there is any space left in the subject, and if there was, it incremented the class size by one and returned the status value ZERO. If there was no space, it did not change the class size; instead it returned the status value –1. (This is a primitive registration system—it just keeps counts!)

As part of the upgrade, *subject_name* has been changed to a two-component structure:

```
structure subject
    string subject_name
    lock slock
```

and the registrar is now wondering where to apply the locking primitives,

```
ACQUIRE (subject.slock)
RELEASE (subject.slock)
```

Here is a typical application program, which registers the caller for two humanities

摩托罗拉68000处理器。当第一个处理器遇到缺页异常时，它会在当前状态直接停止运行，等待操作数出现。第二个摩托罗拉68000处理器（其程序经过精心规划，完全驻留在主内存中）则负责获取缺失的页面，然后重新启动第一个处理器。

其他与摩托罗拉68000合作的设计师采用了一种不同且略显冒险的技巧：修改所有编译器和汇编器，使其仅生成恰好为原子操作的指令。摩托罗拉后来推出了68000的一个版本，其中微处理器的所有内部状态寄存器均可被保存，这一方法与通用电气600系列添加虚拟内存时所采用的相同。

练习

9.1 *Locking up humanities*：教务处的办公室正在升级其针对有限招生人文学科的排课程序。计划将其改为多线程运行，但有人担心多个线程同时尝试更新数据库可能会引发问题。该程序最初仅包含两项操作：

```
status ← 注册 (subject_name)
删除 (subject_name)
```

其中*subject_name* 是一个类似“21W471”的字符串。REGISTER过程会检查该科目是否还有剩余空间，若有，则将班级规模增加一，并返回状态值ZERO。若无剩余空间，则不改变班级规模，而是返回状态值-1。（这是一个简易的注册系统——仅作人数统计！）

作为升级的一部分，*subject_name* 已改为双组件结构：

```
结构 subject   字符串
subject_name
    锁定 slock
```

而注册员现在正琢磨着该在哪里应用这些锁定原语，

```
获取（subject.slock）
发布 (subject.slock)
```

这是一个典型的应用程序，它将调用者注册为两个人文领域的{v*}

subjects, *hx* and *hy*:

```
procedure REGISTER_TWO (hx, hy)
    status ← REGISTER (hx)
    if status = 0 then
        status ← REGISTER (hy)
        if status = −1 then
            DROP (hx)
    return status;
```

9.1a. The goal is that the entire procedure REGISTER_TWO should have the before-or-after property. Add calls for ACQUIRE and RELEASE to the REGISTER_TWO procedure that obey the *simple locking protocol*.

9.1b. Add calls to ACQUIRE and RELEASE that obey the *two-phase locking protocol*, and in addition postpone all ACQUIREs as late as possible and do all RELEASEs as early as possible.

Louis Reasoner has come up with a suggestion that he thinks could simplify the job of programmers creating application programs such as REGISTER_TWO. His idea is to revise the two programs REGISTER and DROP by having them do the ACQUIRE and RELEASE internally. That is, the procedure:

```
procedure REGISTER (subject)
    { current code }
    return status
```

would become instead:

```
procedure REGISTER (subject)
    ACQUIRE (subject.slock)
    { current code }
    RELEASE (subject.slock)
    return status
```

9.1c. As usual, Louis has misunderstood some aspect of the problem. Give a brief explanation of what is wrong with this idea.

1995–3–2a…c

9.2 Ben and Alyssa are debating a fine point regarding version history transaction disciplines and would appreciate your help. Ben says that under the mark point transaction discipline, every transaction should call MARK_POINT_ANNOUNCE as soon as possible, or else the discipline won't work. Alyssa claims that everything will come out correct even if no transaction calls MARK_POINT_ANNOUNCE. Who is right?

2006-0-1

9.3 Ben and Alyssa are debating another fine point about the way that the version history transaction discipline bootstraps. The version of NEW_OUTCOME_RECORD given in the text uses TICKET as well as ACQUIRE and RELEASE. Alyssa says this is overkill—it

主题，*hx* 和 *hy*：
procedure REGISTER_TWO (*hx*, *hy*)
status ← 注册 (*hx*) 如果 *status* = 0 则 *status* ← 注册 (*hy*) 如果 *status* = –1 则 删除 (*hx*) 返回 *status*;

9.1a. 目标是整个流程REGISTER_TWO应具备前后顺序属性。向REGISTER_TWO流程中添加对ACQUIRE和RELEASE的调用，这些调用需遵循*simple locking protocol*。

9.1b. 添加对ACQUIRE和RELEASE的调用，这些调用需遵守*two-phase locking protocol*，此外尽可能推迟所有ACQUIRE，并尽早完成所有RELEASE。

路易斯·里森纳提出了一项他认为能简化程序员创建应用程序工作的建议，比如REGISTER_TWO。他的想法是通过让两个程序REGISTER和DROP在内部完成ACQUIRE和RELEASE来对它们进行修订。也就是说，该流程：

过程 注册 (*subject*)
{ 当前代码 } 返回 *status* 将改为：

过程 注册(*subject*)
获取 (*subject.slock*) { 当前代码 } 释放 (*subject.slock*) 返回 *status*

9.1c. 和往常一样，路易斯误解了问题的某个方面。简要解释一下这个想法的问题所在。

1995–3–2a…c

9.2 本和艾丽莎正在就版本历史事务处理的某个细节进行辩论，他们希望得到你的帮助。本认为，在标记点事务处理规则下，每个事务都应尽快调用MARK_POINT_ANNOUNCE，否则该规则将失效。艾丽莎则声称，即使没有事务调用MARK_POINT_ANNOUNCE，一切仍会正确无误。谁是对的?

2006-0-1

9.3 本和艾丽莎正在就版本历史事务规则如何自举的另一个细节展开辩论。文中给出的NEW_OUTCOME_RECORD版本同时使用了TICKET、ACQUIRE和RELEASE。艾丽莎认为这有些过度——

should be possible to correctly coordinate NEW_OUTCOME_RECORD using just ACQUIRE and RELEASE. Modify the pseudocode of Figure 9.30 to create a version of NEW_OUTCOME_RECORD that doesn't need the ticket primitive.

9.4 You have been hired by Many-MIPS corporation to help design a new 32-register RISC processor that is to have six-way multiple instruction issue. Your job is to coordinate the interaction among the six arithmetic-logic units (ALUs) that will be running concurrently. Recalling the discussion of coordination, you realize that the first thing you must do is decide what constitutes "correct" coordination for a multiple-instruction-issue system. Correct coordination for concurrent operations on a database was said to be:

No matter in what order things are actually calculated, the final result is always guaranteed to be one that could have been obtained by some sequential ordering of the concurrent operations.

You have two goals: (1) maximum performance, and (2) not surprising a programmer who wrote a program expecting it to be executed on a single-instruction-issue machine.

Identify the best coordination correctness criterion for your problem.

A. Multiple instruction issue must be restricted to sequences of instructions that have non-overlapping register sets.

B. No matter in what order things are actually calculated, the final result is always guaranteed to be one that could have been obtained by some sequential ordering of the instructions that were issued in parallel.

C. No matter in what order things are actually calculated, the final result is always guaranteed to be the one that would have been obtained by the original ordering of the instructions that were issued in parallel.

D. The final result must be obtained by carrying out the operations in the order specified by the original program.

E. No matter in what order things are actually calculated, the final result is always guaranteed to be one that could have been obtained by some set of instructions carried out sequentially.

F. The six ALUs do not require any coordination.

1997–0–02

9.5 In 1968, IBM introduced the Information Management System (IMS) and it soon became one of the most widely used database management systems in the world. In fact, IMS is still in use today. At the time of introduction IMS used a before-or-after atomicity protocol consisting of the following two rules:

- A transaction may read only data that has been written by previously committed transactions.
- A transaction must acquire a lock for every data item that it will write.

应该可以仅通过ACQUIRE 和RELEASE来正确协调NEW_OUTCOME_RECORD。修改图9.30的伪代码，创建一个不需要票证原语的NEW_OUTCOME_RECORD版本。

9.4　你受雇于Many-MIPS公司，协助设计一款新型32寄存器RISC处理器，该处理器将实现六路多指令发射。你的职责是协调六个并行运行的算术逻辑单元（ALU）之间的交互。回顾关于协调的讨论后，你意识到首要任务是明确多指令发射系统中何为“正确”的协调。对于数据库并发操作的正确协调，曾定义为：

无论实际计算顺序如何，最终结果始终保证与某种并发操作的顺序执行所能得到的结果一致。

你有两个目标：(1) 最大化性能，(2) 不让程序员感到意外——他们编写的程序原本预期在单指令发射机器上执行。

标识为你的问题确定最佳的协调正确性标准　　米

A. 多指令发射必须限制在寄存器集不重叠的指令序列上。 B. 无论实际计算顺序如何，最终结果始终保证与某种并行发射指令的顺序执行结果一致。 C. 无论实际计算顺序如何，最终结果始终保证与并行发射指令的原始顺序执行结果一致。 D. 最终结果必须按照原始程序指定的操作顺序执行得出。 E. 无论实际计算顺序如何，最终结果始终保证与某种顺序执行的指令集结果一致。 F. 六个算术逻辑单元(ALU)无需任何协调。

1997–0–02

9.5　1968年，IBM推出了信息管理系统（IMS），它迅速成为全球使用最广泛的数据库管理系统之一。事实上，IMS至今仍在使用。在推出之初，IMS采用了一种“前或后”原子性协议，该协议包含以下两条规则：

- 一个事务只能读取由先前已提交事务写入的数据。
- 事务必须为它将写入的每个数据项获取一个锁。

Consider the following two transactions, which, for the interleaving shown, both adhere to the protocol:

```
1        BEGIN (t1);              BEGIN (t2)
2        ACQUIRE (y.lock)
3        temp1 ← x
4                                 ACQUIRE (x.lock)
5                                 temp2 ← y
6                                 x ← temp2
7        y ← temp1
8        COMMIT (t1)
9                                 COMMIT (t2)
```

Previously committed transactions had set $x \leftarrow 3$ and $y \leftarrow 4$.

9.5a. After both transactions complete, what are the values of *x* and *y*? In what sense is this answer wrong?

1982–3–3a

9.5b. In the mid-1970's, this flaw was noticed, and the before-or-after atomicity protocol was replaced with a better one, despite a lack of complaints from customers. Explain why customers may not have complained about the flaw.

1982–3–3b

9.6 A system that attempts to make actions all-or-nothing writes the following type of records to a log maintained on non-volatile storage:

• <STARTED *i*>	action *i* starts.
• <*i*, *x*, *old*, *new*>	action *i* writes the value *new* over the value *old* for the variable *x*.
• <COMMITTED *i*>	action *i* commits.
• <ABORTED *i*>	action *i* aborts.
• <CHECKPOINT *i*, *j*,…>	At this checkpoint, actions *i*, *j*,… are pending.

Actions start in numerical order. A crash occurs, and the recovery procedure finds

考虑以下两个事务，对于所示交错执行，两者均遵循协议：

~~3~~	~~开始 (t1)~~	开始 (*t2*)	2 获取 (*y.lock*)
4		获取 (*x.lock*)	
5			*temp2* ← *y*
6			*x* ← *temp2*
7	*y* ← *temp1*	提交 (*t1*)	9 提交 (*t2*)
8			

先前提交的事务已将*x* ←设为3，*y* ←设为4。

9.5a. 在两次交易都完成后，*x* 和 *y* 的值分别是多少？从什么意义上说这个答案是错误的?

1982–3–3a

9.5b. 在20世纪70年代中期，人们注意到了这一缺陷，尽管客户并未提出投诉，但“前后”原子性协议仍被更优方案取代。试解释客户可能未就该缺陷提出抱怨的原因。

1982–3–3b

9.6 一个试图将操作设为全有或全无的系统，会向非易失性存储中维护的日志写入以下类型的记录：

- <开始 *i*> 动作 *i* 开始。
- <*i*, *x*, *old*, *new*> 动作 *i* 将值 *new* 覆盖到变量 *x* 的原值 *old* 上。

- <已提交*i*> 动作 *i* 提交。
- <已中止*i*> 动作 *i* 中止。
- <检查点 *i*, *j*,…> 在此检查点，操作 *i*、*j*…… 待处理。

动作按数字顺序开始。发生崩溃后，恢复程序发现

the following log records starting with the last checkpoint:

```
<CHECKPOINT 17, 51, 52>
<STARTED 53>
<STARTED 54>
<53, y, 5, 6>
<53, x, 5, 9>
<COMMITTED 53>
<54, y, 6, 4>
<STARTED 55>
<55, z, 3, 4>
<ABORTED 17>
<51, q, 1, 9>
<STARTED 56>
<55, y, 4, 3>
<COMMITTED 54>
<55, y, 3, 7>
<COMMITTED 51>
<STARTED 57>
<56, x, 9, 2>
<56, w, 0, 1>
<COMMITTED 56>
<57, u, 2, 1>
********** crash happened here **********
```

9.6a. Assume that the system is using a rollback recovery procedure. How much farther back in the log should the recovery procedure scan?

9.6b. Assume that the system is using a roll-forward recovery procedure. How much farther back in the log should the recovery procedure scan?

9.6c. Which operations mentioned in this part of the log are winners and which are losers?

9.6d. What are the values of *x* and *y* immediately after the recovery procedure finishes? Why?

1994–3–3

9.7 The log of exercise 9.6 contains (perhaps ambiguous) evidence that someone didn't follow coordination rules. What is that evidence?

1994–3–4

9.8 Roll-forward recovery requires writing the commit (or abort) record to the log *before* doing any installs to cell storage. Identify the best reason for this requirement.

A. So that the recovery manager will know what to undo.
B. So that the recovery manager will know what to redo.
C. Because the log is less likely to fail than the cell storage.
D. To minimize the number of disk seeks required.

1994–3–5

以下日志记录从最后一个检查点开始:

<检查点 17, 51, 52> <启动 53> <启动 54> <53, *y*, 5, 6> <53, *x*, 5, 9 > <提交 53> <54, *y*, 6, 4> <启动 55> <55, *z*, 3, 4> <中止 17> <51 , *q*, 1, 9> <启动 56> <55, *y*, 4, 3> <提交 54> <55, *y*, 3, 7> <提交 51 > <启动 57> <56, *x*, 9, 2> <56, *w*, 0, 1> <提交 56> <57, *u*, 2, 1> * ********** 此处发生崩溃 **********

9.6a. 假设系统正在使用回滚恢复程序。恢复程序应在日志中向后扫描多远? 9. 6b. 假设系统正在使用前滚恢复程序。恢复程序应在日志中向后扫描多远? 9.6c . 日志这部分提到的操作中，哪些是胜者，哪些是败者? 9.6d. 恢复程序完成后 ，*x*和*y*的值分别是多少？为什么?

1994–3–3

9.7 练习9.6的日志中包含了（可能含糊的）证据，表明有人未遵守协调规则 。这一证据是什么?

1994–3–4

9.8 前滚恢复要求在将提交（或中止）记录写入日志*before*之前，先完成对单元存储的任何安装操作。请指出这一要求的最佳理由。

A. 这样恢复管理器就能知道需要撤销哪些操作。 B. 这样恢复管理器就能知道需要重做哪些操作。 C. 因为日志比单元存储更不容易失效。 D. 为了最小化所需的磁盘寻道次数。

1994–3–5

9.9 Two-phase locking within transactions ensures that

A. No deadlocks will occur.
B. Results will correspond to some serial execution of the transactions.
C. Resources will be locked for the minimum possible interval.
D. Neither gas nor liquid will escape.
E. Transactions will succeed even if one lock attempt fails.

1997–3–03

9.10 Pat, Diane, and Quincy are having trouble using e-mail to schedule meetings. Pat suggests that they take inspiration from the 2-phase commit protocol.

9.10a. Which of the following protocols most closely resembles 2-phase commit?

```
I.  a. Pat requests everyone's schedule openings.
    b. Everyone replies with a list but does not guarantee to hold all the times available.
    c. Pat inspects the lists and looks for an open time.
        If there is a time,
            Pat chooses a meeting time and sends it to everyone.
        Otherwise
            Pat sends a message canceling the meeting.

II. a–c, as in protocol I.
    d. Everyone, if they received the second message,
            acknowledge receipt.
        Otherwise
            send a message to Pat asking what happened.

III a–c, as in protocol I.
    d. Everyone, if their calendar is still open at the chosen time
            Send Pat an acknowledgment.
        Otherwise
            Send Pat apologies.
    e. Pat collects the acknowledgments. If all are positive
            Send a message to everyone saying the meeting is ON.
        Otherwise
            Send a message to everyone saying the meeting is OFF.
    f. Everyone, if they received the ON/OFF message,
            acknowledge receipt.
        Otherwise
            send a message to Pat asking what happened.

IV. a–f, as in protocol III.
    g. Pat sends a message telling everyone that everyone has confirmed.
    h. Everyone acknowledges the confirmation.
```

9.10b. For the protocol you selected, which step commits the meeting time?

1994–3–7

9.9 事务内的两阶段锁定确保

A. 不会发生死锁。 B. 结果将与事务的某种串行执行相对应。 C. 资源将被锁定在尽可能短的时间间隔内。 D. 气体和液体均不会逸出。 E. 即使一次锁定尝试失败，事务仍会成功。

1997–3–03

9.10 帕特、黛安和昆西在使用电子邮件安排会议时遇到了困难。帕特建议他们从两阶段提交协议中汲取灵感。 9.10a. 以下哪种协议与两阶段提交最为相似?

一、 a. 帕特征求每个人的空闲时间安排。 b. 大家回复各自的空闲时段列表，但无法保证所有时间都能预留。 c. 帕特检查这些列表，寻找共同空闲时间。 若存在合适时段， 帕特选定会议时间并通知所有人。 否则 帕特将发送取消会议的通知。

二、 a–c，同协议一。 d. 所有人若收到第二条消息， 需确认接收。 否则 向Pat发送消息询问情况。

III a–c，同协议I。 d. 每个人，如果在选定时间其日程仍为开放 向Pat发送确认。 否则 向Pat发送致歉。 e. Pat收集确认。若全部为肯定 向所有人发送会议“如期举行”的通知。 否则 向所有人发送会议“取消”的通知。f. 每个人，若收到“举行/取消”通知， 需确认接收。 否则 向Pat发送询问情况的消息。

IV. a–f，同协议III。 g. Pat发送一条消息告知所有人均已确认。 h. 所有人对确认进行回应。 9.10b. 对于您选择的协议，哪一步骤确定了会议时间?

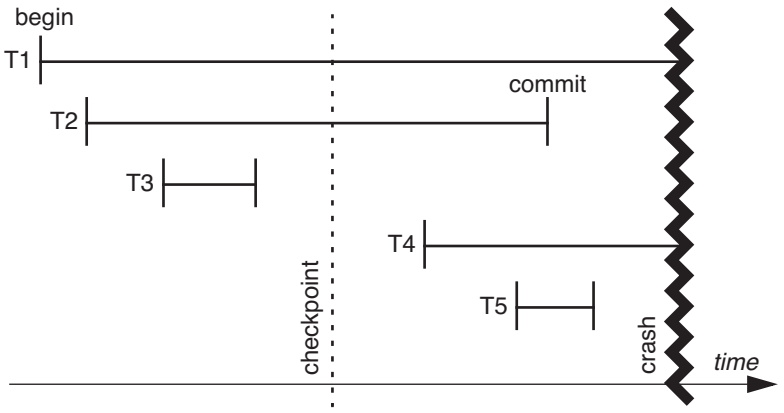
1994–3–7

9.11 Alyssa P. Hacker needs a transaction processing system for updating information about her collection of 97 cockroaches.*

9.11a. In her first design, Alyssa stores the database on disk. When a transaction commits, it simply goes to the disk and writes its changes in place over the old data. What are the major problems with Alyssa's system?

9.11b. *I*n Alyssa's second design, the *only* structure she keeps on disk is a log, with a reference copy of all data in volatile RAM. The log records every change made to the database, along with the transaction which the change was a part of. Commit records, also stored in the log, indicate when a transaction commits. When the system crashes and recovers, it replays the log, redoing each committed transaction, to reconstruct the reference copy in RAM. What are the disadvantages of Alyssa's second design?

To speed things up, Alyssa makes an occasional checkpoint of her database. To checkpoint, Alyssa just writes the entire state of the database into the log. When the system crashes, she starts from the last checkpointed state, and then redoes or undoes some transactions to restore her database. Now consider the five transactions in the illustration:



Transactions T2, T3, and T5 committed before the crash, but T1 and T4 were still pending.

9.11c. When the system recovers, after the checkpointed state is loaded, some transactions will need to be undone or redone using the log. For each transaction,

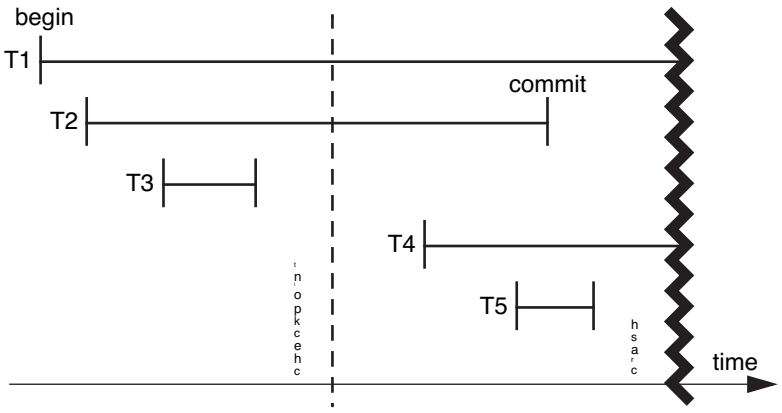
* Credit for developing exercise 9.11 goes to Eddie Kohler.

9.11 Alyssa P. Hacker需要一个事务处理系统来更新她收集的97只蟑螂的相关信息。*

9.11a. 在艾丽莎的第一个设计中，她将数据库存储在磁盘上。当一个事务提交时，它直接访问磁盘，并在旧数据的位置上就地写入更改。艾丽莎的系统存在哪些主要问题?

9.11b. 在阿丽莎的第二个设计方案中，她存储在磁盘上的*only*结构是一个日志，其中包含易失性RAM中所有数据的参考副本。该日志记录了数据库的每一次更改，以及更改所属的事务。同样存储在日志中的提交记录，用于标识事务何时提交。当系统崩溃并恢复时，它会重放日志，重新执行每个已提交的事务，以重建RAM中的参考副本。阿丽莎的第二个设计有哪些缺点?

为了加快进程，艾丽莎会偶尔对她的数据库进行一次检查点操作。进行检查点时，艾丽莎只需将数据库的完整状态写入日志。当系统崩溃时，她从最后一个检查点状态开始，然后重做或撤销某些事务以恢复数据库。现在考虑图示中的五个事务：



事务T2、T3和T5在崩溃前已提交，但T1和T4仍处于待定状态。

9.11c. 当系统恢复时，在加载检查点状态后，需要使用日志对某些事务进行撤销或重做。对于每个事务，

* Credit for developing exercise 9.11 goes to Eddie Kohler.

mark off in the table whether that transaction needs to be undone, redone, or neither.

	Undone	Redone	Neither
T1			
T2			
T3			
T4			
T5			

9.11d. Now, assume that transactions T2 and T3 were actually *nested* transactions: T2 was nested in T1, and T3 was nested in T2. Again, fill in the table

	Undone	Redone	Neither
T1			
T2			
T3			
T4			
T5			

1996–3–3

9.12 Alice is acting as the coordinator for Bob and Charles in a two-phase commit protocol. Here is a log of the messages that pass among them:

1 Alice ⇒ Bob: please do X
2 Alice ⇒ Charles: please do Y
3 Bob ⇒ Alice: done with X
4 Charles ⇒ Alice: done with Y
5 Alice ⇒ Bob: PREPARE to commit or abort
6 Alice ⇒ Charles: PREPARE to commit or abort
7 Bob ⇒ Alice: PREPARED
8 Charles ⇒ Alice: PREPARED
9 Alice ⇒ Bob: COMMIT
10 Alice ⇒ Charles: COMMIT

At which points in this sequence is it OK for Bob to abort his part of the

在表格中标记该事务是需要撤销、重做，还是两者都不需要。

	Undone	Redone	Neither
T1			
T2			
T3			
T4			
T5			

9.11d. 现在，假设事务T2和T3实际上是*nested*事务：T2嵌套在T1中，而T3又嵌套在T2内。再次填写表格

	Undone	Redone	Neither
T1			
T2			
T3			
T4			
T5			

1996–3–3

9.12 Alice 在 两阶段提交协议 中 担任 Bob 和 Charles 的协调者。以下是他们之间传递的消息日志:

1 爱丽丝 ⇒ 鲍勃：请执行X *2* 爱丽丝 ⇒ 查尔斯：请执行Y *3* 鲍勃 ⇒ 爱丽丝：X已完成 *4* 查尔斯 ⇒ 爱丽丝：Y已完成 *5* 爱丽丝 ⇒ 鲍勃：准备提交或中止 *6* 爱丽丝 ⇒ 查尔斯：准备提交或中止 *7* 鲍勃 ⇒ 爱丽丝：已准备就绪 *8* 查尔斯 ⇒ 爱丽丝：已准备就绪 *9* 爱丽丝 ⇒ 鲍勃：提交 *10* 爱丽丝 ⇒ 查尔斯：提交

在这个序列的哪些节点上，Bob可以中止他的部分操作

transaction?

A. After Bob receives message 1 but before he sends message 3.
B. After Bob sends message 3 but before he receives message 5.
C. After Bob receives message 5 but before he sends message 7.
D. After Bob sends message 7 but before he receives message 9.
E. After Bob receives message 9.

2008–3–11

Additional exercises relating to Chapter 9 can be found in problem sets *29* through *40*.

交易?

A. 在Bob收到消息1之后，但在发送消息3之前。 B. 在Bob发送消息3之后，但在收到消息5之前。 C. 在Bob收到消息5之后，但在发送消息7之前。 D. 在Bob发送消息7之后，但在收到消息9之前。 E. 在Bob收到消息9之后。

2008–3–11

与第9章相关的额外练习题可以在习题集***29*** 至***40***中找到。

Glossary for Chapter 9

abort—Upon deciding that an all-or-nothing action cannot or should not commit, to undo all of the changes previously made by that all-or-nothing action. After aborting, the state of the system, as viewed by anyone above the layer that implements the all-or-nothing action, is as if the all-or-nothing action never existed. Compare with *commit*. [Ch. 9]

all-or-nothing atomicity—A property of a multistep action that if an anticipated failure occurs during the steps of the action, the effect of the action from the point of view of its invoker is either never to have started or else to have been accomplished completely. Compare with *before-or-after atomicity* and *atomic*. [Ch. 9]

archive—A record, usually kept in the form of a log, of old data values, for auditing, recovery from application mistakes, or historical interest. [Ch. 9]

atomic (adj.); **atomicity** (n.)—A property of a multistep action that there be no evidence that it is composite above the layer that implements it. An atomic action can be before-or-after, which means that its effect is as if it occurred either completely before or completely after any other before-or-after action. An atomic action can also be all-or-nothing, which means that if an anticipated failure occurs during the action, the effect of the action as seen by higher layers is either never to have started or else to have completed successfully. An atomic action that is *both* all-or-nothing and before-or-after is known as a *transaction*. [Ch. 9]

atomic storage—Cell storage for which a multicell PUT can have only two possible outcomes: (1) it stores all data successfully, or (2) it does not change the previous data at all. In consequence, either a concurrent thread or (following a failure) a later thread doing a GET will always read either all old data or all new data. Computer architectures in which multicell PUTs are not atomic are said to be subject to *write tearing*. [Ch. 9]

before-or-after atomicity—A property of concurrent actions: Concurrent actions are before-or-after actions if their effect from the point of view of their invokers is the same as if the actions occurred either completely before or completely after one another. One consequence is that concurrent before-or-after software actions cannot discover the composite nature of one another (that is, one action cannot tell that another has multiple steps). A consequence in the case of hardware is that concurrent before-or-after WRITEs to the same memory cell will be performed in some order, so there is no danger that the cell will end up containing, for example, the OR of several WRITE values. The database literature uses the words "isolation" and "serializable", the operating system literature uses the words "mutual exclusion" and "critical section", and the computer architecture literature uses the unqualified word "atomicity" for this concept. Compare with *all-or-nothing atomicity* and *atomic*. [Ch. 9]

blind write—An update to a data value *X* by a transaction that did not previously read *X*. [Ch. 9]

第9章术语表

中止（abort）——在判定一个全有或全无（all-or-nothing）动作无法或不应提交时，撤销该动作之前所做的所有更改。中止后，从实现全有或全无动作的层次之上观察，系统的状态如同该动作从未存在过。与{v*}对比。[第9章]

全有或全无原子性——一种多步骤操作的属性，指如果在操作步骤中发生预期故障，从调用者的角度看，该操作的效果要么从未开始，要么已完全完成。与*before-or-after atomicity*和*atomic*进行比较。[第9章]

存档（archive）——一种通常以日志形式保存的记录，用于存储旧数据值，目的包括审计、从应用程序错误中恢复或满足历史研究需求。[第9章]

原子性（adj.）；原子性（n.）——一种多步操作的特性，即在实现该操作的层次之上无迹象表明其具有复合性。原子操作可以是"前或后"的，这意味着其效果表现为要么完全在其他任何"前或后"操作之前发生，要么完全在其之后发生。原子操作也可以是"全有或全无"的，这意味着若操作期间发生预期故障，高层所见的操作效果要么从未开始，要么已成功完成。一个同时满足*both*"全有或全无"和"前或后"特性的原子操作被称为*transaction*。[第9章]

原子存储——一种多单元PUT存储，其仅可能产生两种结果：(1) 成功存储所有数据，或 (2) 完全不改变原有数据。因此，无论是并发线程还是（在发生故障后）后续执行GET的线程，读取到的数据要么全部为旧数据，要么全部为新数据。若计算机架构中的多单元PUT操作不具备原子性，则称其易受*write tearing*影响。[第9章]

前后原子性——并发操作的一种属性：若从调用者的视角看，并发操作的效果等同于这些操作要么完全在另一个之前执行，要么完全在另一个之后执行，则这些操作具有前后原子性。一个必然结果是，具有前后原子性的并发软件操作无法察觉彼此的复合性质（即一个操作无法识别另一个操作包含多个步骤）。在硬件层面的体现是，对同一内存单元的并发前后WRITE操作将按某种顺序执行，因此不存在该单元最终存储例如多个WRITE值的OR的风险。数据库文献使用"隔离性"和"可串行化"表述此概念，操作系统文献采用"互斥"和"临界区"，而计算机体系结构文献则直接使用"原子性"这一术语。对比*all-or-nothing atomicity*与*atomic*。[第9章]

盲写——事务对数据值*X*的更新，而该事务之前并未读取*X*。[第9章]

cell storage—Storage in which a WRITE or PUT operates by overwriting, thus destroying previously stored information. Many physical storage devices, including magnetic disk and CMOS random access memory, implement cell storage. Compare with *journal storage*. [Ch. 9]

checkpoint—1. (n.) Information written to non-volatile storage that is intended to speed up recovery from a crash. 2 (v.) To write a checkpoint. [Ch. 9]

close-to-open consistency—A consistency model for file operations. When a thread opens a file and performs several write operations, all of the modifications weill be visible to concurrent threads only after the first thread closes the file. [Ch. 4]

coheerence—See *read/write coherence* or *cache coherence*.

commit—To renounce the ability to abandon an all-or-nothing action unilaterally. One usually commits an all-or-nothing action before making its results available to concurrent or later all-or-nothing actions. Before committing, the all-or-nothing action can be abandoned and one can pretend that it had never been undertaken. After committing, the all-or-nothing action must be able to complete. A committed all-or-nothing action cannot be abandoned; if it can be determined precisely how far its results have propagated, it may be possible to reverse some or all of its effects by compensation. Commitment also usually includes an expectation that the results preserve any appropriate invariants and will be durable to the extent that the application requires those properties. Compare with *compensate* and *abort*. [Ch. 9]

compensate (adj.); **compensation** (n.)—To perform an action that reverses the effect of some previously committed action. Compensation is intrinsically application dependent; it is easier to reverse an incorrect accounting entry than it is to undrill an unwanted hole. [Ch. 9]

do action—(n.) Term used in some systems for a *redo action*. [Ch. 9]

exponential random backoff—A form of *exponential backoff* in which an action that repeatedly encounters interference repeatedly doubles (or, more generally, multiplies by a constant greater than one) the size of an interval from which it randomly chooses its next delay before retrying. The intent is that by randomly changing the timing relative to other, interfering actions, the interference will not recur. [Ch. 9]

force—(v.) When output may be buffered, to ensure that a previous output value has actually been written to durable storage or sent as a message. Caches that are not write-through usually have a feature that allows the invoker to force some or all of their contents to the secondary storage medium. [Ch. 9]

install—In a system that uses logs to achieve all-or-nothing atomicity, to write data to cell storage. [Ch. 9]

journal storage—Storage in which a WRITE or PUT appends a new value, rather than overwriting a previously stored value. Compare with *cell storage*. [Ch. 9]

lock point—In a system that provides before-or-after atomicity by locking, the first instant in a before-or-after action when every lock that will ever be in its lock set has been

单元存储——一种存储方式，其中WRITE或PUT通过覆写操作运行，从而破坏先前存储的信息。包括磁盘和CMOS随机存取存储器在内的许多物理存储设备都实现了单元存储。与*journal storage*进行比较。[第9章]

检查点—1. (名词) 写入非易失性存储器的信息，旨在加速系统崩溃后的恢复。2 (动词) 执行检查点写入操作。[第9章]

关闭-打开一致性（close-to-open consistency）——一种针对文件操作的并发一致性模型。当某个线程打开文件并执行若干写入操作时，所有修改内容只有在首个线程关闭文件后，才会对其他并发线程可见。[第4章]

一致性——参见*read/write coherence*或*cache coherence*。 提交——指放弃单方面中止一项全有或全无操作的能力。通常，在将全有或全无操作的结果提供给并发或后续的全有或全无操作之前，会先提交该操作。提交前，全有或全无操作可被中止，且可当作其从未执行过；提交后，该操作必须能够完成。已提交的全有或全无操作不可中止；若能精确确定其结果的传播范围，则可能通过补偿来逆转部分或全部效果。提交通常还包含一种预期，即结果将保持所有适当的不变性，并根据应用需求确保这些属性的持久性。对比*compensate*和*abort*。[第9章]

补偿（形容词）；补偿（名词）——执行一个动作以逆转之前某个已执行动作的效果。补偿本质上依赖于具体应用；相比撤销一个不想要的钻孔，纠正一个错误的会计条目要容易得多。[第9章]

执行动作——（名词）在某些系统中用于指代*redo action*的术语。[第9章]

指数随机退避——一种*exponential backoff*形式，其中反复遭遇干扰的动作会不断将延迟区间的大小加倍（或更一般地说，乘以大于1的常数），然后从中随机选择下一次重试前的等待时间。其目的是通过相对于其他干扰动作随机调整时机，避免干扰再次发生。[第9章]

强制（动词）——当输出可能被缓冲时，确保先前的输出值已实际写入持久存储或作为消息发送。非直写式缓存通常具备一项功能，允许调用者强制将其部分或全部内容推送至二级存储介质。[第9章]

安装——在使用日志实现全有或全无原子性的系统中，将数据写入单元存储。[第9章]

日志存储——一种存储方式，其中WRITE或PUT会追加新值，而非覆盖先前存储的值。与*cell storage*进行比较。[第9章]

锁定点——在通过锁定提供前后原子性的系统中，一个前后动作中的第一个瞬间，此时其锁定集中将包含的所有锁都已被获取。

acquired. [Ch. 9]

lock set—The collection of all locks acquired during the execution of a before-or-after action. [Ch. 9]

log—1. (n.) A specialized use of journal storage to maintain an append-only record of some application activity. Logs are used to implement all-or-nothing actions, for performance enhancement, for archiving, and for reconciliation. 2. (v.) To append a record to a log. [Ch. 9]

logical locking—Locking of higher-layer data objects such as records or fields of a database. Compare with *physical locking*. [Ch. 9]

mark point—1. (adj.) An atomicity-assuring discipline in which each newly created action *n* must wait to begin reading shared data objects until action (*n – 1*) has marked all of the variables it intends to modify. 2. (n.) The instant at which an action has marked all of the variables it intends to modify. [Ch. 9]

optimistic concurrency control—A concurrency control scheme that allows concurrent threads to proceed even though a risk exists that they will interfere with each other, with the plan of detecting whether there actually is interference and, if necessary, forcing one of the threads to abort and retry. Optimistic concurrency control is an effective technique in situations where interference is possible but not likely. Compare with *pessimistic concurrency control*. [Ch. 9]

page fault—See *missing-page exception*.

pair-and-spare—See *pair-and-compare*.

pending—A state of an all-or-nothing action, when that action has not yet either committed or aborted. Also used to describe the value of a variable that was set or changed by a still-pending all-or-nothing action. [Ch. 9]

pessimistic concurrency control—A concurrency control scheme that forces a thread to wait if there is any chance that by proceeding it may interfere with another, concurrent, thread. Pessimistic concurrency control is an effective technique in situations where interference between concurrent threads has a high probability. Compare with *optimistic concurrency control*. [Ch. 9]

physical locking—Locking of lower-layer data objects, typically chunks of data whose extent is determined by the physical layout of a storage medium. Examples of such chunks are disk sectors or even an entire disk. Compare with *logical locking*. [Ch. 9]

prepaging—An optimization for a multilevel memory manager in which the manager predicts which pages might be needed and brings them into the primary memory before the application demands them. Compare with *demand algorithm*.

prepared—In a layered or multiple-site all-or-nothing action, a state of a component action that has announced that it can, on command, either commit or abort. Having reached this state, it awaits a decision from the higher-layer coordinator of the action. [Ch. 9]

presented load—See *offered load*.

获取。[第9章]　锁集合——在执行一个“之前或之后”动作期间获取的所有锁的集合。[第9章]　日志—1.（名词）一种专门用途的日志存储，用于维护某些应用活动的仅追加记录。日志用于实现全有或全无动作、性能提升、归档及对账。2.（动词）向日志追加一条记录。[第9章]　逻辑锁定——对更高层数据对象（如数据库记录或字段）的锁定。与*physical locking*比较。[第9章]　标记点——1.（形容词）一种确保原子性的规则，其中每个新创建的动作n必须等待，直到动作$(n-1)$标记了它打算修改的所有变量后，才能开始读取共享数据对象。2.（名词）动作标记完其打算修改的所有变量的瞬间。[第9章]　乐观并发控制——一种并发控制方案，允许并发线程继续执行，即使存在它们可能相互干扰的风险，计划是检测是否实际发生干扰，并在必要时强制其中一个线程中止并重试。在干扰可能但不太可能发生的情况下，乐观并发控制是一种有效的技术。与*pessimistic concurrency control*比较。[第9章]

页面错误——参见*missing-page exception*。

配对与备用——参见*pair-and-compare*。　待定pending）——全有或全无操作的一种状态，表示该操作尚未提交或中止。也用于描述由仍处于待定状态的全有或全无操作所设置或更改的变量值。[第9章]　悲观并发控制——一种并发控制方案，当线程有可能干扰其他并发线程时，强制其等待。在并发线程间干扰概率较高的场景中，悲观并发控制是一种有效技术。对比*optimistic concurrency control*。[第9章]　物理锁定——对底层数据对象的锁定，通常锁定由存储介质物理布局决定的数据块。例如磁盘扇区或整个磁盘。对比*logical locking*。[第9章]　预分页——多级内存管理器的优化策略，管理器预测可能需要的页面并在应用程序请求前将其预加载至主存。对比*demand algorithm*。

已准备prepared）——在分层或多站点的全有或全无操作中，组件操作的一种状态，表明其已声明可根据指令提交或中止。进入此状态后，等待上层协调器作出决策。[第9章]　呈现负载——参见*offered load*。

progress—A desirable guarantee provided by an atomicity-assuring mechanism: that despite potential interference from concurrency some useful work will be done. An example of such a guarantee is that the atomicity-assuring mechanism will not abort at least one member of the set of concurrent actions. In practice, lack of a progress guarantee can sometimes be repaired by using exponential random backoff. In formal analysis of systems, progress is one component of a property known as "liveness". Progress is an assurance that the system will move toward some specified goal, whereas liveness is an assurance that the system will eventually reach that goal. [Ch. 9]

redo action—An application-specified action that, when executed during failure recovery, produces the effect of some committed component action whose effect may have been lost in the failure. (Some systems call this a "do action". Compare with *undo action*.) [Ch. 9]

roll-forward recovery—A write-ahead log protocol with the additional requirement that the application log its outcome record *before* it performs any install actions. If there is a failure before the all-or-nothing action passes its commit point, the recovery procedure does not need to undo anything; if there is a failure after commit, the recovery procedure can use the log record to ensure that cell storage installs are not lost. Also known as *redo logging*. Compare with *rollback recovery*. [Ch. 9]

rollback recovery—A write-ahead log protocol with the additional requirement that the application perform all install actions *before* logging an outcome record. If there is a failure before the all-or-nothing action commits, a recovery procedure can use the log record to undo the partially completed all-or-nothing action. Also known as *undo logging*. Compare with *roll-forward recovery*. [Ch. 9]

serializable—A property of before-or-after actions, that even if several operate concurrently, the result is the same as if they had acted one at a time, in some sequential (in other words, serial) order. [Ch. 9]

shadow copy—A working copy of an object that an all-or-nothing action creates so that it can make several changes to the object while the original remains unmodified. When the all-or-nothing action has made all of the changes, it then carefully exchanges the working copy with the original, thus preserving the appearance that all of the changes occurred atomically. Depending on the implementation, either the original or the working copy may be identified as the "shadow" copy, but the technique is the same in either case. [Ch. 9]

simple locking—A locking protocol for creating before-or-after actions requiring that no data be read or written before reaching the lock point. For the atomic action to also be all-or-nothing, a further requirement is that no locks be released before commit (or abort). Compare with *two-phase locking*. [Ch. 9]

simple serialization—An atomicity protocol requiring that each newly created atomic action must wait to begin execution until all previously started atomic actions are no longer pending. [Ch. 9]

transaction—A multistep action that is both atomic in the face of failure and atomic in the

进展（progress）——由原子性保障机制提供的一种理想保证：尽管存在并发带来的潜在干扰，系统仍能完成某些有用工作。此类保证的一个例子是，原子性保障机制至少不会中止并发操作集中的任一成员。实践中，缺乏进展保证时，有时可通过指数随机退避机制进行补救。在系统的形式化分析中，进展是称为“活性”（liveness）属性的组成部分。进展确保系统将朝着特定目标推进，而活性则保证系统最终会达成该目标。[第9章]

redo action 重做操作（redo action）——应用程序指定的一个操作，在故障恢复期间执行时，能产生某些已提交组件操作的效果（这些效果可能在故障中丢失）。（某些系统称之为“执行操作”。对比*undo action.*）[第9章]

roll-forward recovery 前滚恢复（roll-forward recovery）——一种预写日志协议，额外要求应用程序在执行任何安装操作前记录其结果记录*before*。若全有或全无动作在通过提交点前发生故障，恢复过程无需撤销任何操作；若在提交后发生故障，恢复过程可利用日志记录确保存储单元的安装操作不会丢失。亦称*redo logging*。对比*rollback recovery*。[第9章]

rollback recovery 回滚恢复（rollback recovery）——一种预写日志协议，额外要求应用程序在记录结果记录前执行所有安装操作*before*。若全有或全无动作提交前发生故障，恢复过程可利用日志记录撤销部分完成的动作。亦称*undo logging*。对比*roll-forward recovery*。[第9章]

serializable 可串行化（serializable）——先于或后于操作的属性，即使多个操作并发执行，其结果仍等同于它们按某种顺序（即串行）依次执行的效果。[第9章]

shadow copy 影子副本（shadow copy）——全有或全无动作为对象创建的工作副本，用于在保持原对象不变的同时进行多次修改。当完成所有修改后，该动作会谨慎地将工作副本与原对象交换，从而保持所有修改原子性发生的表象。根据实现方式，原对象或工作副本可能被标识为“影子”副本，但两种情况下技术原理相同。[第9章]

simple locking 简单锁（simple locking）——创建先于或后于操作的锁协议，要求在到达锁定点前不得读取或写入任何数据。若要使原子动作同时满足全有或全无性，还需满足在提交（或中止）前不得释放任何锁。对比*two-phase locking*。[第9章]

simple serialization 简单串行化（simple serialization）——一种原子性协议，要求每个新创建的原子动作必须等待所有先前启动的原子动作不再处于待处理状态后才能开始执行。[第9章]

transaction 事务（transaction）——一种多步骤动作，既具备故障原子性，又具备...

face of concurrency. That is, it is both all-or-nothing and before-or-after. [Ch. 9]

transactional memory—A memory model in which multiple references to primary memory are both all-or-nothing and before-or-after. [Ch. 9]

two generals dilemma—An intrinsic problem that no finite protocol can guarantee to simultaneously coordinate state values at two places that are linked by an unreliable communication network. [Ch. 9]

two-phase commit—A protocol that creates a higher-layer transaction out of separate, lower-layer transactions. The protocol first goes through a preparation (sometimes called voting) phase, at the end of which each lower-layer transaction reports either that it cannot perform its part or that it is prepared to either commit or abort. It then enters a commitment phase in which the higher-layer transaction, acting as a coordinator, makes a final decision—thus the name two-phase. Two-phase commit has no connection with the similar-sounding term *two-phase locking*. [Ch. 9]

two-phase locking—A locking protocol for before-or-after atomicity that requires that no locks be released until all locks have been acquired (that is, there must be a lock point). For the atomic action to also be all-or-nothing, a further requirement is that no locks for objects to be written be released until the action commits. Compare with *simple locking*. Two-phase locking has no connection with the similar-sounding term *two-phase commit*. [Ch. 9]

undo action—An application-specified action that, when executed during failure recovery or an abort procedure, reverses the effect of some previously performed, but not yet committed, component action. The goal is that neither the original action nor its reversal be visible above the layer that implements the action. Compare with *redo* and *compensate*. [Ch. 9]

version history—The set of all values for an object or variable that have ever existed, stored in journal storage. [Ch. 9]

write-ahead-log (WAL) protocol—A recovery protocol that requires appending a log record in journal storage before installing the corresponding data in cell storage. [Ch. 9]

write tearing—See *atomic storage*.

并发性的体现。也就是说，它既是全有或全无的，也是之前或之后的。[第9章] 事务性内存——一种内存模型，其中对主内存的多次引用既是全有或全无的，也是之前或之后的。[第9章]

两将军问题——一个本质性问题，即任何有限协议都无法保证在由不可靠通信网络连接的两个地点间同步协调状态值{v*}。[第9章]

两阶段提交——一种将独立的底层事务组合成高层事务的协议。该协议首先经历一个准备阶段（有时称为投票阶段），在此阶段结束时，每个底层事务会报告其无法执行自身部分，或已准备好提交或中止。随后进入提交阶段，由高层事务作为协调者做出最终决定——因此得名“两阶段”。两阶段提交与发音相似的术语*two-phase locking*无任何关联。[第9章]

两阶段锁定——一种用于实现“之前或之后”原子性的锁定协议，要求在所有锁被获取之前不得释放任何锁（即必须存在一个锁定点）。为了确保原子操作同时也是“全有或全无”的，进一步要求是：在操作提交之前，不得释放任何待写入对象的锁。与*simple locking*进行比较。两阶段锁定与听起来相似的术语*two-phase commit*并无关联。[第9章]

撤销操作——一种由应用程序指定的动作，在故障恢复或中止过程中执行时，能够逆转某些先前执行但尚未提交的组件操作的效果。其目的是使原始操作及其逆转在实现该操作的层次之上均不可见。与*redo*和*compensate*进行比较。[第9章]

版本历史——一个对象或变量曾经存在过的所有值的集合，存储在日志存储中。[第9章]　预写日志W(AL）协议——一种恢复协议，要求在将相应数据安装到单元存储之前，先在日志存储中追加一条日志记录。[第9章]　写撕裂——参见*atomic storage*。

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第9章索引

设计原则和提示出现在*underlined italics*中。过程名称出现在SMALL CAPS中。粗体页码位于章节词汇表内。