Concurrency Control

Some Slides taken from Mohammad Hammoud (CMU, Qatar).

Schedules

- A schedule is a list of actions (i.e., read, write, abort, and/or commit) from a *set* of transactions

- The *order* in which two actions of a transaction ***T*** appear in a schedule must be the same as they appear in ***T*** itself

- Assume **T1** = [R(A), W(A)] and **T2** = [R(B), W(B), R(C), W(C)]

T1	**T2**
R(A)	
	R(B)
W(A)	
	W(B)
	R(C)
	W(C)

✓

T1	**T2**
R(A)	
W(A)	
	R(B)
	W(B)
	R(C)
	W(C)

✓

T1	**T2**
R(A)	
W(A)	
	R(C)
	W(C)
	R(B)
	W(B)

✗

Serial Schedules

- A complete schedule must contain all the actions of every transaction that appears on it

- If the actions of different transactions are ***not interleaved***, the schedule is called a serial schedule

T1	T2
R(A)	
W(A)	
Commit	
	R(A)
	W(A)
	R(C)
	W(C)
	Commit

A Serial Schedule

T1	T2
	R(B)
	W(B)
R(A)	
W(A)	
Commit	
	R(C)
	W(C)
	Commit

A Non-Serial Schedule

Anomalies

- Interleaving actions of different transactions can leave the database in an inconsistent state

- Two actions on the same data object are said to ***conflict*** if at least one of them is a write

- There are 3 anomalies that can arise upon interleaving actions of different transactions (say, T1 and T2):
  - Write-Read (WR) Conflict: T2 reads a data object previously written by T1
  - Read-Write (RW) Conflict: T2 writes a data object previously read by T1
  - Write-Write (WW) Conflict: T2 writes a data object previously written by T1

Implementation of Isolation Levels

- Locking
  - Lock on whole database vs lock on items
  - How long to hold lock?
  - Shared vs exclusive locks
- Timestamps
  - Transaction timestamp assigned e.g. when a transaction begins
  - Data items store two timestamps
    - Read timestamp
    - Write timestamp
  - Timestamps are used to detect out of order accesses
- Multiple versions of each data item
  - Allow transactions to read from a “snapshot” of the database

Lock-Based Protocols

- WR, RW and WW anomalies can be avoided using a *locking protocol*

- A lock is a mechanism to control concurrent access to a data item
- Data items can be locked in two modes :

1. **exclusive** *(X) mode*. Data item can be both read as well as write. X-lock is requested using **lock-X** instruction.
2. **shared** *(S) mode*. Data item can only be read. S-lock is requested using **lock-S** instruction.

- Lock requests are made to concurrency-control manager. Transaction can proceed only after request is granted.
- The part of the DBMS that keeps track of locks is called the *lock manager*

Lock-Based Protocols (Cont.)

- **Lock-compatibility matrix**

	S	X
S	true	false
X	false	false

- A transaction may be granted a lock on an item if the requested lock is compatible with locks already held on the item by other transactions
- Any number of transactions can hold shared locks on an item,
- But if any transaction holds an exclusive on the item, no other transaction may hold any lock on the item.

Lock-Based Protocols (Cont.)

- Example of a transaction performing locking:

```
T_2: lock-S(A);
     read (A);
     unlock(A);

     lock-S(B);
     read (B);
     unlock(B);
     display(A+B)
```

- Locking as above is *not sufficient* to guarantee serializability

Schedule With Lock Grants

- Grants omitted in rest of chapter
  - Assume grant happens just before the next instruction following lock request
- This schedule is not serializable (why?)
- A **locking protocol** is a set of rules followed by all transactions while requesting and releasing locks.
- Locking protocols enforce serializability by restricting the set of possible schedules.

T_1	T_2	concurrency-control manager
lock-X(B)		
		grant-X(B, T_1)
read(B)		
$B := B - 50$		
write(B)		
unlock(B)		
	lock-S(A)	
		grant-S(A, T_2)
	read(A)	
	unlock(A)	
	lock-S(B)	
		grant-S(B, T_2)
	read(B)	
	unlock(B)	
	display($A + B$)	
lock-X(A)		
		grant-X(A, T_1)
read(A)		
$A := A + 50$		
write(A)		
unlock(A)		

Deadlock

- Locking can lead to an undesirable situation. For example, consider the partial schedule

T_3	T_4
lock-X(B)	
read(B)	
$B := B - 50$	
write(B)	
	lock-S(A)
	read(A)
	lock-S(B)
lock-X(A)	

- Neither T_3 nor T_4 can make progress — executing **lock-S**(*B*) causes T_4 to wait for T_3 to release its lock on *B*, while executing **lock-X**(*A*) causes T_3 to wait for T_4 to release its lock on *A*.
- Such a situation is called a **deadlock**.
  - To handle a deadlock one of T_3 or T_4 must be rolled back and its locks released.

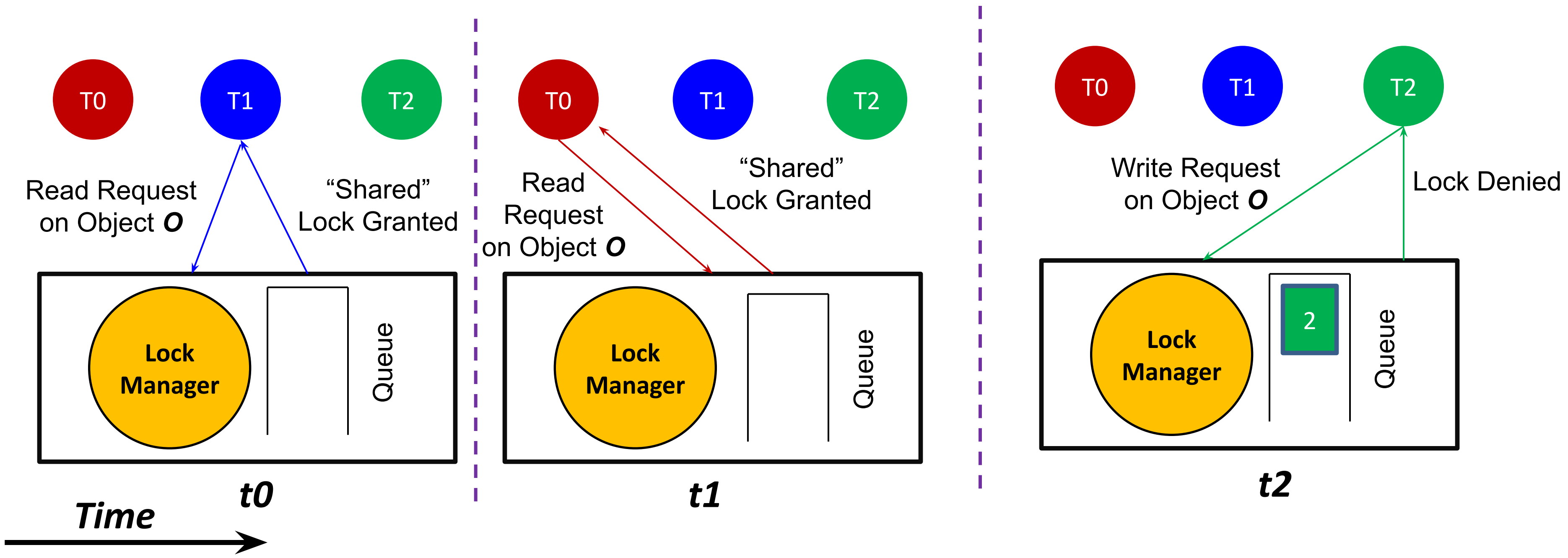
Deadlock (Cont.)

- The potential for deadlock exists in most locking protocols. Deadlocks are a necessary evil.
- **Starvation** is also possible if concurrency control manager is badly designed. For example:
  - A transaction may be waiting for an X-lock on an item, while a sequence of other transactions request and are granted an S-lock on the same item.
  - The same transaction is repeatedly rolled back due to deadlocks.
- Concurrency control manager can be designed to prevent starvation.

- Starvation of transactions can be avoided by granting the locks as follows:
  - When T_i requests a lock on a data item *Q* in a particular mode *M*, the concurrency control manager grants the lock provided that
    - *There is no other transaction holding a lock on Q in a mode that conflicts with M*
    - *There is no other transaction that is waiting for a lock on Q and that made its lock request before* T_i

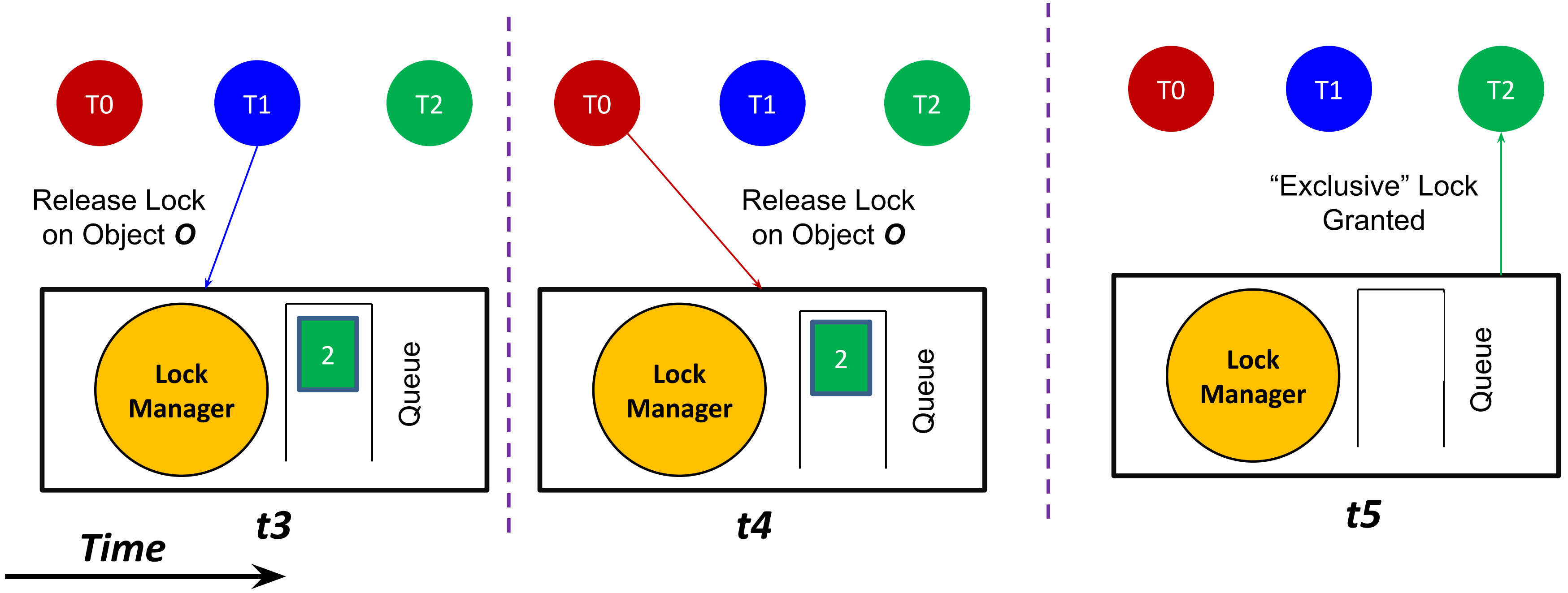
Two-Phase Locking

- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
  - Rule 1: if a transaction ***T*** wants to read (or write) an object ***O***, it first requests the lock manager for a shared (or exclusive) lock on ***O***



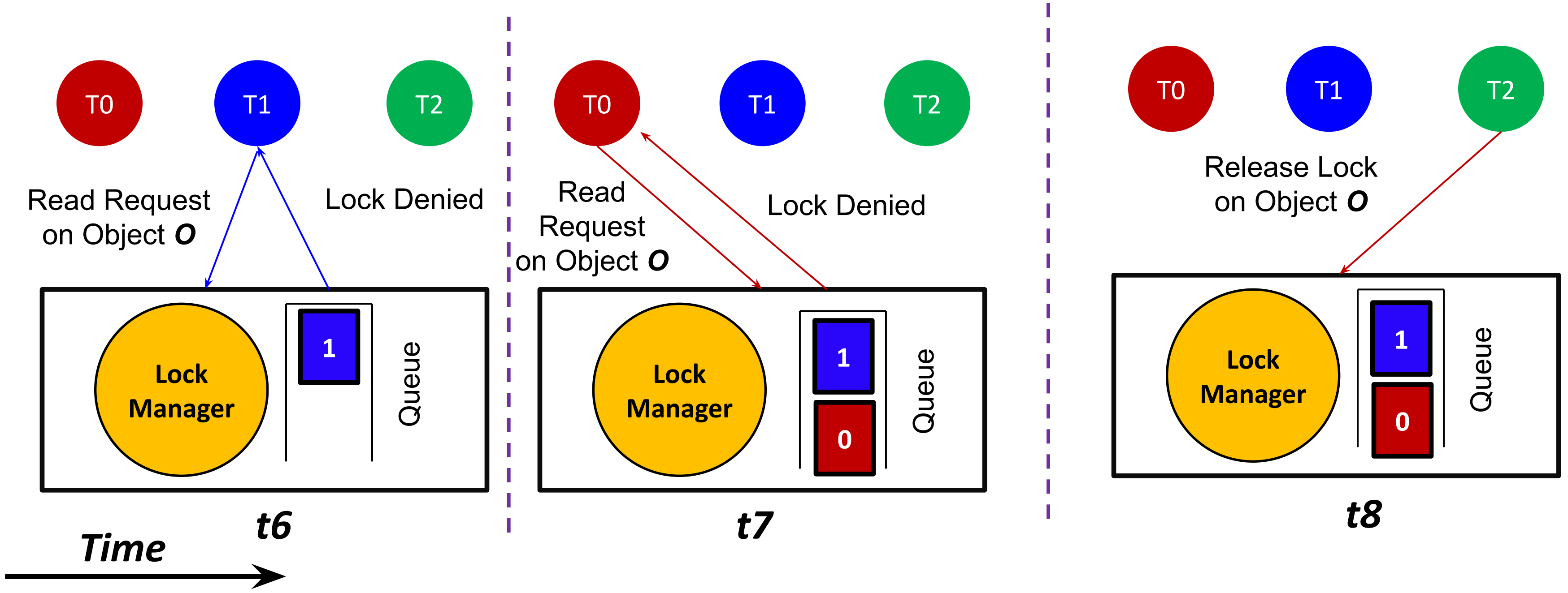
Two-Phase Locking

- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
  - Rule 1: if a transaction ***T*** wants to read (or write) an object ***O***, it first requests the lock manager for a shared (or exclusive) lock on ***O***



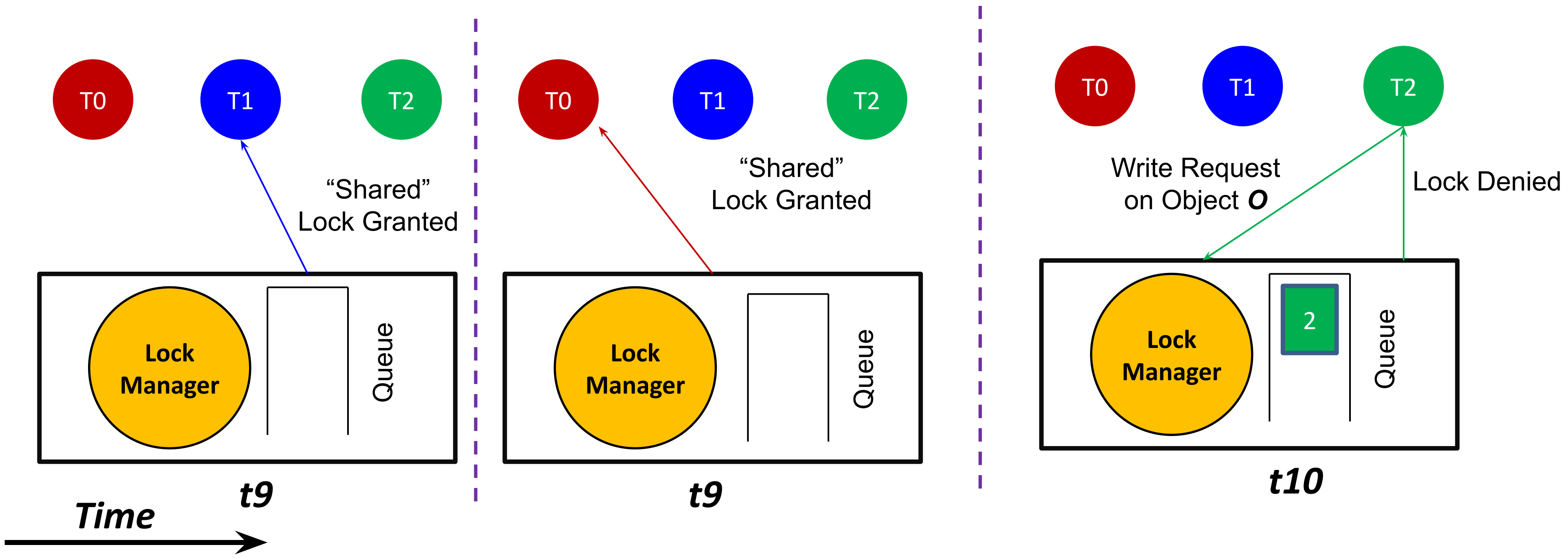
Two-Phase Locking

- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
  - Rule 1: if a transaction ***T*** wants to read (or write) an object ***O***, it first requests the lock manager for a shared (or exclusive) lock on ***O***



Two-Phase Locking

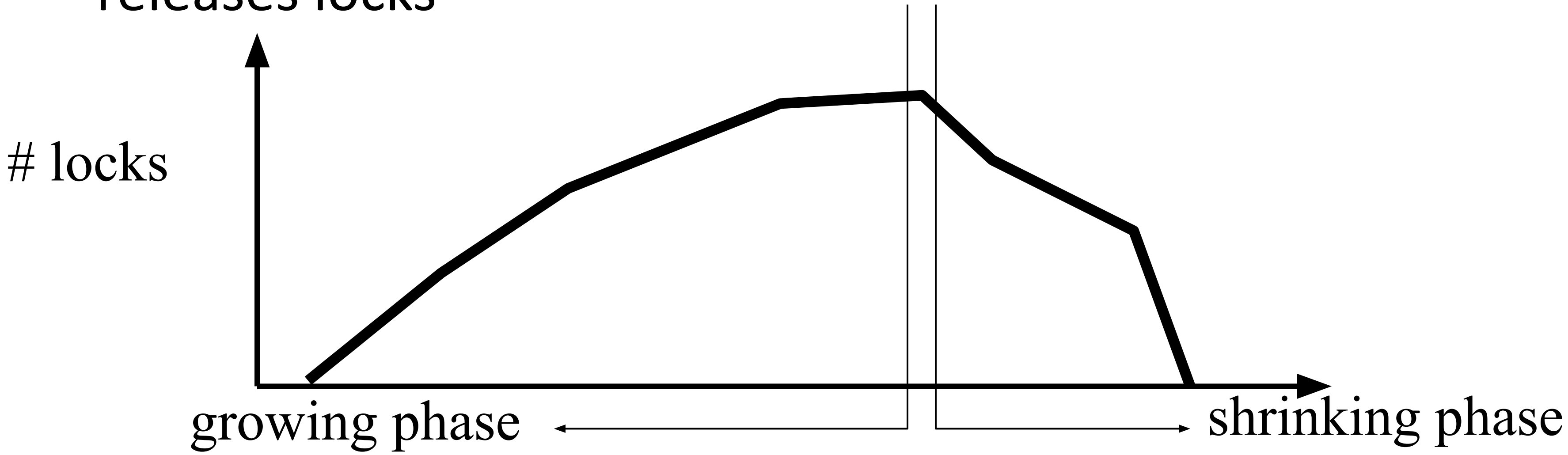
- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
    - Rule 1: if a transaction ***T*** wants to read (or write) an object ***O***, it first requests the lock manager for a shared (or exclusive) lock on ***O***



Two-Phase Locking

- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
  - Rule 2: ***T*** can release locks before it *commits* or *aborts*, and cannot request additional locks once it releases *<u>any</u>* lock

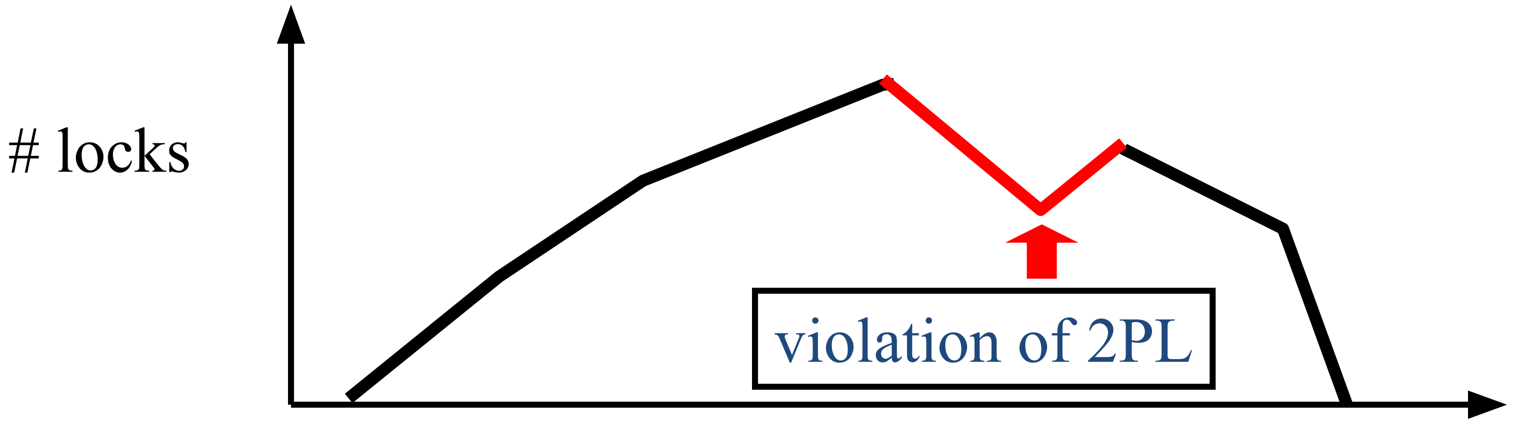
- Thus, every transaction has a "growing" phase in which it acquires locks, followed by a "shrinking" phase in which it releases locks



Two-Phase Locking

- A widely used locking protocol, called *Two-Phase Locking* (*2PL*), has two rules:
  - Rule 2: ***T*** can release locks before it *commits* or *aborts*, and cannot request additional locks once it releases *any* lock

- Thus, every transaction has a “growing” phase in which it acquires locks, followed by a “shrinking” phase in which it releases locks



Resolving RW Conflicts Using 2PL

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 reads A
  - T2 reads A, decrements A and commits
  - T1 tries to decrement A

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
	R(A)
	W(A)
	Commit
W(A)	
Commit	

Exposes RW Anomaly

T1	**T2**
EXCLUSIVE(A)	
R(A)	
	Lock(A)
	$\uparrow$ **Wait** $\downarrow$
W(A)	
Commit	
	EXCLUSIVE(A)
	R(A)
	W(A)
	Commit

RW Conflict Resolved!

Resolving RW Conflicts Using 2PL

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 reads A
  - T2 reads A, decrements A and commits
  - T1 tries to decrement A

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
	R(A)
	W(A)
	Commit
W(A)	
Commit	

Exposes RW Anomaly

T1	**T2**
EXCLUSIVE(A)	
R(A)	
	Lock(A)
W(A)	
	↑ **Wait** ↓
Commit	
	EXCLUSIVE(A)
	R(A)
	W(A)
	Commit

But, it can limit parallelism!

Resolving WW Conflicts Using 2PL

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 sets Mohammad's Salary to $1000
  - T2 sets Ahmad's Salary to $2000
  - T1 sets Ahmad's Salary to $1000
  - T2 sets Mohammad's Salary to $2000

- T1 and T2 can be represented by the following schedule:

T1	**T2**
W(MS)	
	W(AS)
W(AS)	
Commit	W(MS)
	Commit

Exposes WW Anomaly
(*assuming, MS & AS must be kept equal*)

T1	**T2**
EXCLUSIVE(MS)	
EXCLUSIVE(AS)	
W(MS)	**Lock**(AS)
W(AS)	↑ **Wait** ↓
Commit	
	EXCLUSIVE(AS)
	EXCLUSIVE(MS)
	W(AS)
	W(MS)
	Commit

WW Conflict Resolved!

Resolving WW Conflicts Using 2PL

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 sets Mohammad's Salary to $1000
  - T2 sets Ahmad's Salary to $2000
  - T1 sets Ahmad's Salary to $1000
  - T2 sets Mohammad's Salary to $2000
- T1 and T2 can be represented by the following schedule:

T1	**T2**
W(MS)	
	W(AS)
W(AS)	
Commit	
	W(MS)
	Commit

Exposes WW Anomaly
(*assuming, MS & AS must be kept equal*)

T1	**T2**
EXCLUSIVE(MS)	
W(MS)	
	EXCLUSIVE(AS)
Lock(AS)	W(AS)
↑	**Lock**(MS)
Wait	↑
	Wait

Deadlock!

Cascading rollback may occur under two-phase locking

T5	T6	T7
lock-X(A)		
read(A)		
lock-S(B)		
read(B)		
write(A)		
unlock(A)		
	lock-X(A)	
	read(A)	
	write(A)	
	unlock(A)	
		lock-S(A)
		read(A)
abort		

Cascading rollback may occur under two-phase locking.

Resolving WR Conflicts

- Suppose that T1 and T2 actions are *interleaved* as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B
  - T1 credits $100 to account B

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
R(B)	
W(B)	
Commit	

Exposes WR Anomaly

T1	**T2**
EXCLUSIVE(A)	
EXCLUSIVE(B)	**Lock**(A)
R(A)	**Lock**(B)
W(A)	↑
R(B)	**Wait**
W(B)	↓
Commit	
	EXCLUSIVE(A)
	EXCLUSIVE(B)
	R(A)
	W(A)
	R(B)
	W(B)
	Commit

WR Conflict Resolved!

Resolving WR Conflicts

- Suppose that T1 and T2 actions are *interleaved* as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B
  - T1 credits $100 to account B
- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
R(B)	
W(B)	
Commit	

Exposes WR Anomaly

T1	**T2**
EXCLUSIVE(A)	
EXCLUSIVE(B)	**Lock**(A)
R(A)	**Lock**(B)
W(A)	**Wait**
RELEASE(A)	↓
R(B)	**EXCLUSIVE**(A)
W(B)	R(A)
Commit	W(A)
	EXCLUSIVE(B)
	R(B)
	W(B)
	Commit

WR Conflict is ***NOT*** Resolved!

How can we solve this?

Strict Two-Phase Locking

- WR conflicts (as well as RW & WW) can be solved by making 2PL *stricter*

- In particular, *Rule 2* in 2PL can be modified as follows:
  - Rule 2: locks of a transaction ***T*** can only be released after ***T*** completes (i.e., commits or aborts)

- This version of 2PL is called *Strict* Two-Phase Locking

Resolving WR Conflicts: *Revisit*

- Suppose that T1 and T2 actions are *interleaved* as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B
  - T1 credits $100 to account B

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
R(B)	
W(B)	
Commit	

Exposes WR Anomaly

T1	**T2**
EXCLUSIVE(A)	
EXCLUSIVE(B)	**Lock**(A)
R(A)	**Lock**(B)
W(A)	**Wait**
RELEASE(A)	↓
R(B)	**EXCLUSIVE**(A)
W(B)	R(A)
Commit	W(A)
	EXCLUSIVE(B)
	R(B)
	W(B)
	Commit

Not allowed with *strict* 2PL

Resolving WR Conflicts: *Revisit*

- Suppose that T1 and T2 actions are *interleaved* as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B
  - T1 credits $100 to account B

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
R(B)	
W(B)	
Commit	

Exposes WR Anomaly

T1	**T2**
EXCLUSIVE(A)	
EXCLUSIVE(B)	**Lock**(A)
R(A)	**Lock**(B)
W(A)	↑
R(B)	**Wait**
W(B)	↓
Commit	
	EXCLUSIVE(A)
	EXCLUSIVE(B)
	R(A)
	W(A)
	R(B)
	W(B)
	Commit

WR Conflict is Resolved!

But, parallelism is limited more!

2PL vs. Strict 2PL

- Two-Phase Locking (2PL):
  - Limits concurrency
  - May lead to deadlocks
  - May have 'dirty reads'

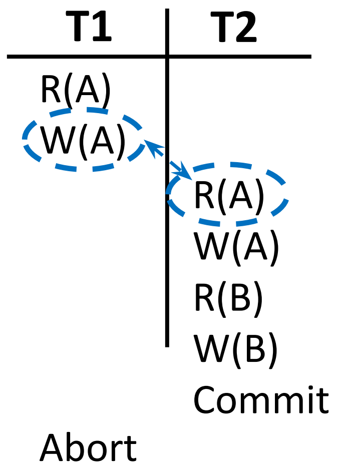
- Strict 2PL:
  - Limits concurrency more (*but*, actions of different transactions can still be interleaved)
  - May still lead to deadlocks
  - Avoids 'dirty reads'

T1	**T2**
SHARED(A)	
R(A)	
	SHARED(A)
	R(A)
	EXECLUSIVE(B)
	R(B)
EXCLUSIVE(C)	W(B)
R(C)	Commit
W(C)	
Commit	

A Schedule with *Strict 2PL* and *Interleaved Actions*

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted
- T1 and T2 can be represented by the following schedule:



T1	T2
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

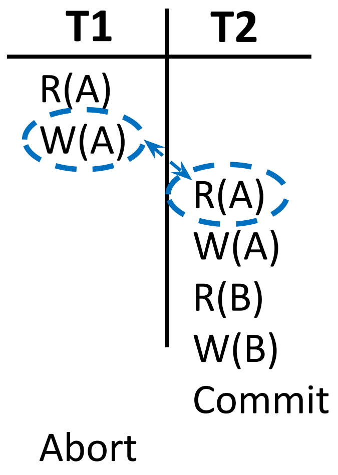
T2 read a value for A that should have never been there!

How can we deal with the situation, assuming T2 *had not yet committed*?

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted

- T1 and T2 can be represented by the following schedule:



T1	T2
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

T2 read a value for A that should have never been there!

We can *cascade* the abort of T1 by aborting T2 as well!

This "cascading process" can be *recursively* applied to any transaction that read A written by T1

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted

- T1 and T2 can be represented by the following schedule:

T1	T2
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

T2 read a value for A that should have never been there!

How can we deal with the situation, assuming T2 *had actually committed*?

The schedule is indeed *unrecoverable*!

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

T2 read a value for A that should have never been there!

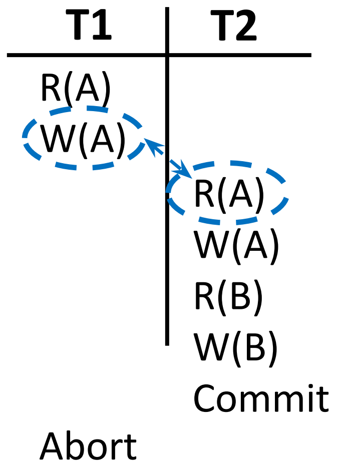
For a schedule to be *recoverable*, transactions should commit only after all transactions whose changes they read commit!

"Recoverable schedules" avoid *cascading aborts*!

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted

- T1 and T2 can be represented by the following schedule:



T1	T2
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

T2 read a value for A that should have never been there!

How can we ensure "recoverable schedules"?

By using Strict 2PL!

Schedules with *Aborted* Transactions

- Suppose that T1 and T2 actions are interleaved as follows:
  - T1 deducts $100 from account A
  - T2 adds 6% interest to accounts A and B, and commits
  - T1 is aborted

- T1 and T2 can be represented by the following schedule:

T1	**T2**
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

T1	**T2**
EXCLUSIVE(A)	
R(A)	
W(A)	**Lock**(A)
	Wait
Abort	↓
UNDO(T1)	**EXCLUSIVE**(A)
	R(A)
	W(A)
	EXCLUSIVE(B)
	R(B)
	W(B)
	Commit

Cascaded aborts are avoided!

Serializable Schedules: *Redefined*

- Two schedules are said to be *equivalent* if for any database state, the effect of executing the 1st schedule is <u>identical</u> to the effect of executing the 2nd schedule

- <u>Previously</u>: a *serializable schedule* is a schedule that is equivalent to a serial schedule

- <u>Now</u>: a *serializable schedule* is a schedule that is equivalent to a serial schedule <u>*over a set of committed transactions*</u>

- This definition captures *serializability* as well as *recoverability*

Lock Conversions

- A transaction may need to change the lock it already acquires on an object
  - From Shared to Exclusive
    - This is referred to as *lock upgrade*
  - From Exclusive to Shared
    - This is referred to as *lock downgrade*

- This protocol ensures serializability

- For example, an SQL update statement might acquire a Shared lock *on each row*, ***R***, in a table and if ***R*** satisfies the condition (in the WHERE clause), an Exclusive lock must be obtained for ***R***

Lock Upgrades

- A lock upgrade request from a transaction ***T*** on object ***O*** must be handled specially by:
  - Granting an Exclusive lock to ***T*** immediately *if no other transaction holds a lock on* **O**
  - Otherwise, queuing ***T*** at the *front* of ***O***'s queue (i.e., *T is favored*)

- ***T*** is *favored* because it already holds a Shared lock on ***O***
  - Queuing ***T*** *in front of* another transaction ***T'*** that holds no lock on ***O***, but requested an Exclusive lock on ***O*** averts a deadlock!
  - However, if ***T*** and ***T'*** hold a Shared lock on ***O***, and both request upgrades to an Exclusive lock, a deadlock will arise regardless!

Lock Downgrades

- Lock upgrades can be entirely avoided by obtaining Exclusive locks *initially*, and downgrade them to Shared locks once needed

- Would this violate any 2PL requirement?
  - On the surface yes; since the transaction (say, ***T***) may need to upgrade later

  - This is a special case as ***T*** *<u>conservatively</u>* obtained an Exclusive lock, and did nothing but read the object that it downgraded

  - 2PL can be safely extended to allow lock downgrades in the growing phase, *<u>provided that the transaction has not modified the object</u>*

Automatic Acquisition of Locks

- A transaction T_i issues the standard read/write instruction, without explicit locking calls.
- The operation **read**(Q) is processed as:

```
          if no lock on Q
            then
                 grant T_i a  lock-S on Q;
      read(Q)
            else begin
                   if necessary wait until no other
                      transaction has a lock-X on Q
                   grant T_i a  lock-S on Q;
                   read(Q)
                 end
```

Automatic Acquisition of Locks (Cont.)

- The operation **write**(*Q*) is processed as:

```
if no lock on Q
  then
  grant $T_i$ a lock-X on Q
  write(Q)
 else begin
     if necessary wait until no other trans. has any lock on Q,
     if $T_i$ has a lock-S on Q
         then
           upgrade lock on Q  to lock-X
        else
           grant $T_i$ a lock-X on Q
        write(Q)
   end;
```

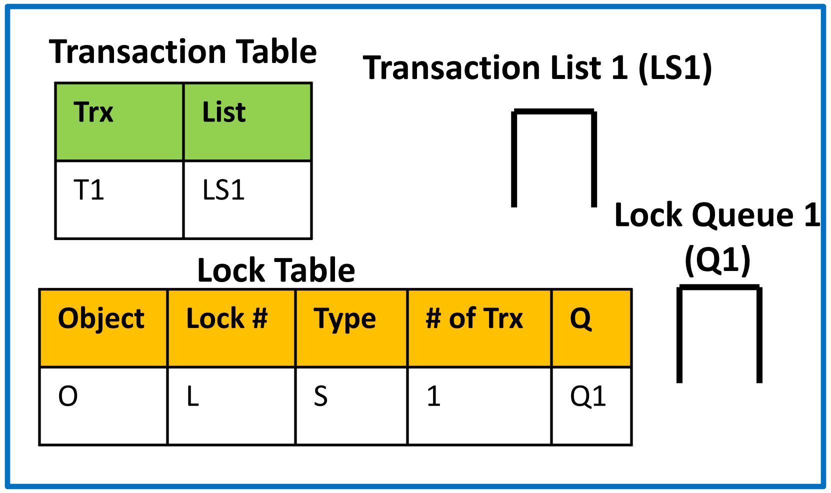
- **All locks are released after commit or abort**

Implementation of Locking

- A **lock manager** can be implemented as a separate process
- Transactions can send lock and unlock requests as messages
- The lock manager replies to a lock request by sending a lock grant messages (or a message asking the transaction to roll back, in case of a deadlock)
  - The requesting transaction waits until its request is answered
- The lock manager maintains an in-memory data-structure called a **lock table** to record granted locks and pending requests

Lock Manager Implementation

- Usually, a lock manager in a DBMS maintains three types of data structures:
  - A queue, ***Q***, for each lock, ***L***, to hold its pending requests
  - A lock table, which keeps for each ***L*** associated with each object, ***O***, a record ***R*** that contains:
    - The type of ***L*** (e.g., shared or exclusive)
    - The number of transactions currently holding ***L*** on ***O***
    - A pointer to ***Q***
  - A transaction table, which maintains for each transaction, ***T***, a pointer to a list of locks held by ***T***



Transaction Table

Trx	**List**
T1	LS1

Lock Table

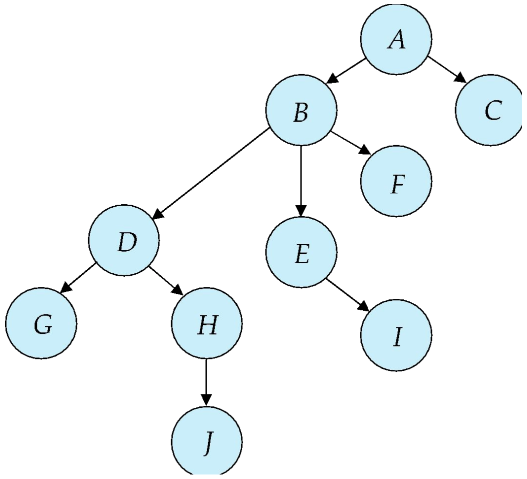
Object	**Lock #**	**Type**	**# of Trx**	**Q**
O	L	S	1	Q1

Graph-Based Protocols

- Graph-based protocols are an alternative to two-phase locking
- Impose a partial ordering $\rightarrow$ on the set **D** = $\{d_1, d_2, \ldots, d_h\}$ of all data items.
  - If $d_i \rightarrow d_j$ then any transaction accessing both d_i and d_j must access d_i before accessing d_j.
  - Implies that the set **D** may now be viewed as a directed acyclic graph, called a *database graph*.
- The *tree-protocol* is a simple kind of graph protocol.

Tree Protocol

- Only exclusive locks are allowed.
- The first lock by T_i may be on any data item. Subsequently, a data Q can be locked by T_i only if the parent of Q is currently locked by T_i.
- Data items may be unlocked at any time.
- A data item that has been locked and unlocked by T_i cannot subsequently be relocked by T_i



Graph-Based Protocols (Cont.)

- The tree protocol ensures conflict serializability as well as freedom from deadlock.
- Unlocking may occur earlier in the tree-locking protocol than in the two-phase locking protocol.
  - Shorter waiting times, and increase in concurrency
  - Protocol is deadlock-free, no rollbacks are required
- Drawbacks
  - Protocol does not guarantee recoverability or cascade freedom
    - Need to introduce commit dependencies to ensure recoverability
  - Transactions may have to lock data items that they do not access.
    - increased locking overhead, and additional waiting time
    - potential decrease in concurrency
- Schedules not possible under two-phase locking are possible under the tree protocol, and vice versa.

Deadlock Handling

- System is **deadlocked** if there is a set of transactions such that every transaction in the set is waiting for another transaction in the set.

T_3	T_4
lock-X(B)	
read(B)	
$B := B - 50$	
write(B)	
	lock-S(A)
	read(A)
	lock-S(B)
lock-X(A)	

Deadlock Detection

- The lock manager maintains a structure called a *waits-for graph* to *periodically* detect deadlocks

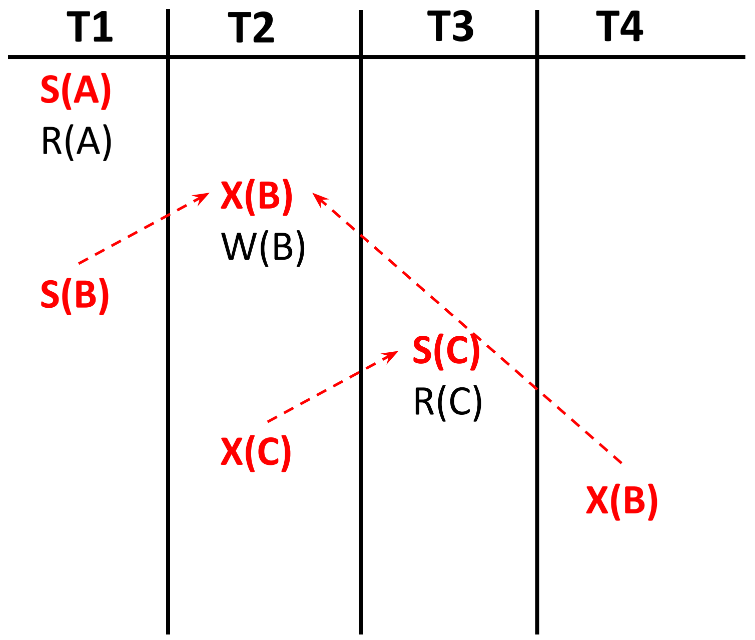
- In a waits-for graph:
  - The nodes correspond to active transactions
  - There is an edge from Ti to Tj *if and only if* Ti is waiting for Tj to release a lock

- The lock manager *adds* and *removes* edges to and from a waits-for graph when it *queues* and *grants* lock requests, respectively

- A deadlock is detected when a *cycle* in the waits-for graph is found

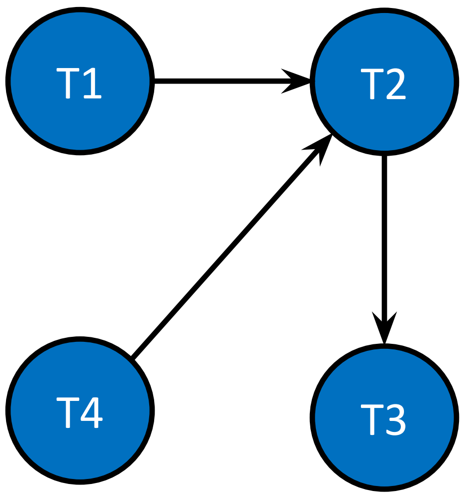
Deadlock Detection (*Cont'd*)

- The following schedule is free of deadlocks:



T1	T2	T3	T4
S(A)			
R(A)			
	X(B)		
	W(B)		
S(B)			
		S(C)	
		R(C)	
	X(C)		
			X(B)

A schedule *without* a deadlock



$$\begin{tikzcd} T1 \arrow[r] & T2 \arrow[d] \\ T4 \arrow[ur] & T3 \end{tikzcd}$$

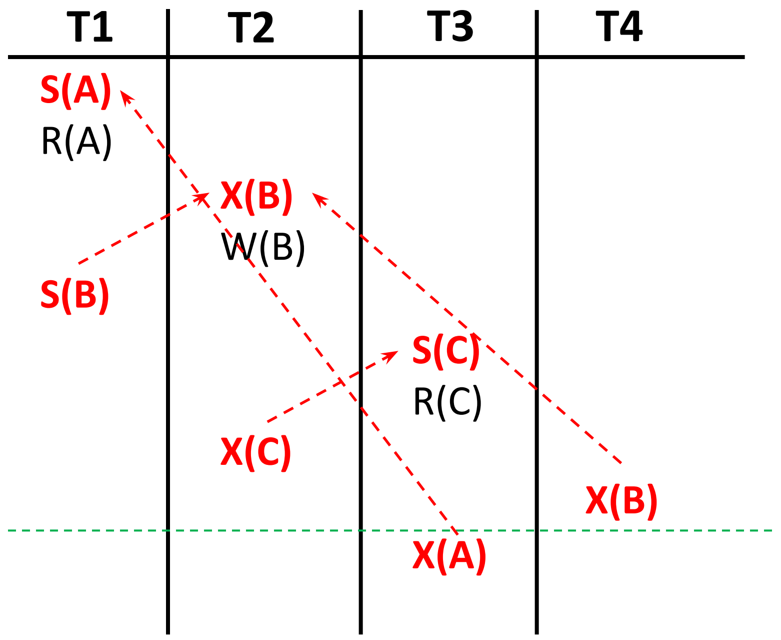
No cycles; hence, no deadlocks!

The Corresponding **Waits-For Graph***

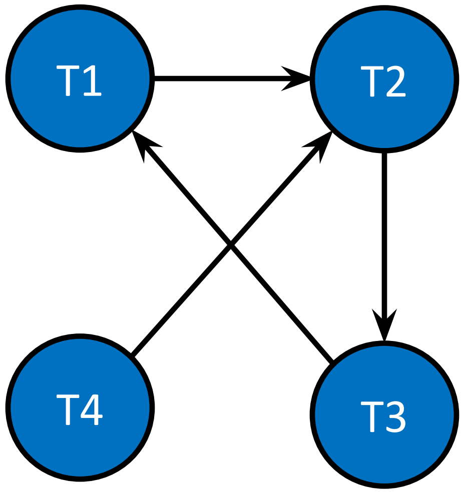
*The nodes correspond to active transactions and there is an edge from Ti to Tj *if and only if* Ti is waiting for Tj to release a lock

Deadlock Detection (*Cont'd*)

- The following schedule is **NOT** free of deadlocks:



A schedule *with* a deadlock

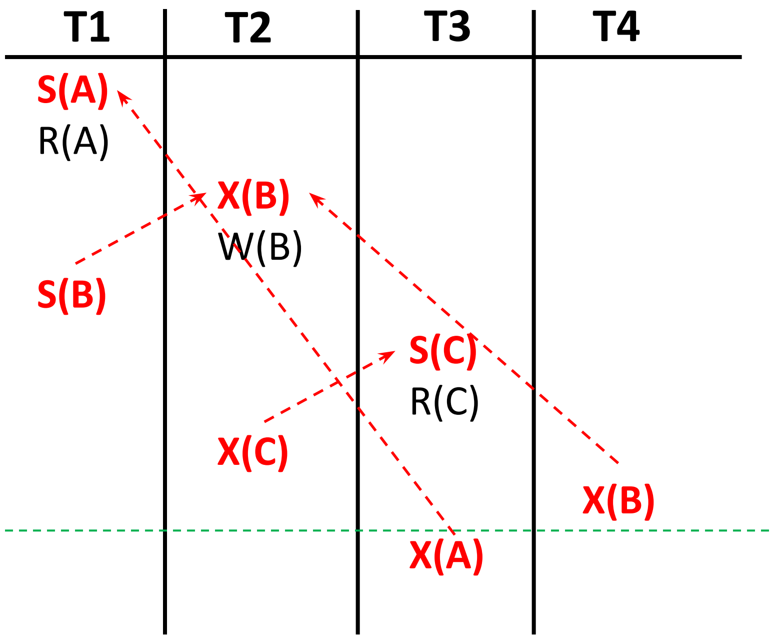


The Corresponding **Waits-For Graph***

*The nodes correspond to active transactions and there is an edge from Ti to Tj *if and only if* Ti is waiting for Tj to release a lock

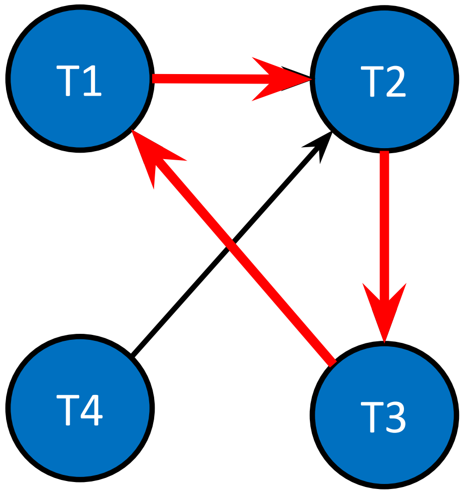
Deadlock Detection (*Cont'd*)

- The following schedule is **NOT** free of deadlocks:



T1	T2	T3	T4
S(A)			
R(A)			
	X(B)		
	W(B)		
S(B)			
		S(C)	
		R(C)	
	X(C)		
			X(B)
		X(A)	

A schedule *with* a deadlock



Cycle detected; hence, a deadlock!

The Corresponding **Waits-For Graph***

*The nodes correspond to active transactions and there is an edge from Ti to Tj *if and only if* Ti is waiting for Tj to release a lock

Resolving Deadlocks

- A deadlock is resolved by aborting a transaction that is on a cycle and releasing its locks
  - This allows some of the waiting transactions to proceed

- When deadlock is detected :
  - Some transaction will have to rolled back (made a **victim**) to break deadlock cycle.
    - Select that transaction as victim that will incur minimum cost
  - Rollback -- determine how far to roll back transaction
    - **Total rollback**: Abort the transaction and then restart it.
    - **Partial rollback**: Roll back victim transaction only as far as necessary to release locks that another transaction in cycle is waiting for

- The choice of which transaction to abort can be made using different criteria:
  - The one with the fewest locks
  - Or the one that has done the least work
  - Or the one that is farthest from completion (*more accurate*)

- Starvation can happen (why?)
  - One solution: oldest transaction in the deadlock set is never chosen as victim

  - Caveat: a transaction that was aborted in the past, should be *favored* subsequently and not aborted upon a deadlock detection!

Deadlock Prevention

- Studies indicate that deadlocks are relatively infrequent and *detection-based schemes* work well in practice
- However, if there is a high level of *contention* for locks, *prevention-based schemes* could perform better
- Deadlocks can be averted by giving each transaction a *priority* and ensuring that lower-priority transactions are not allowed to wait for higher-priority ones
  (or vice versa)

Deadlock Prevention Strategies

- One way to assign priorities is to give each transaction a *timestamp* when it is started
  - Thus, the lower the timestamp, the higher is the transaction's priority, i.e. older the transaction.
- If a transaction ***Ti*** requests a lock and a transaction ***Tj*** holds a conflicting lock, the lock manager can use one of the following policies:
  - Wait-Die scheme (non-preemptive): **If *Ti* is older than *Tj, Ti* is allowed to wait; otherwise, Tj is rolled back (dies).** For example, T1, T2, and T3 have timestamps 5, 10, and 15, respectively. If T1 requests a data item held by T2, then T1 will wait, If T3 requests a data item held by T2, then T3 will be rolled back.
    - Older transaction may wait for younger one to release data item.
    - Younger transactions never wait for older ones; they are rolled back instead.
    - A transaction may die several times before acquiring a lock
  - Wound-Wait scheme(preemptive): **If *Ti* is older than *Tj, Tj* is rolled back (i.e. Tj is wounded by Ti); otherwise, Tj is allowed to wait.** For example, T1, T2, and T3 have timestamps 5, 10, and 15, respectively. If T1 requests a data item held by T2, then T2 will be rolled back. If T3 requests a data item held by T2, then T3 will wait.
    - Older transaction *wounds* of younger transaction instead of waiting for it.
    - Younger transactions may wait for older ones.
    - Fewer rollbacks than *wait-die* scheme.
- In both schemes, a rolled back transactions is restarted with its original timestamp.
  - Ensures that older transactions have precedence over newer ones, and starvation is thus avoided.

Deadlock prevention (Cont.)

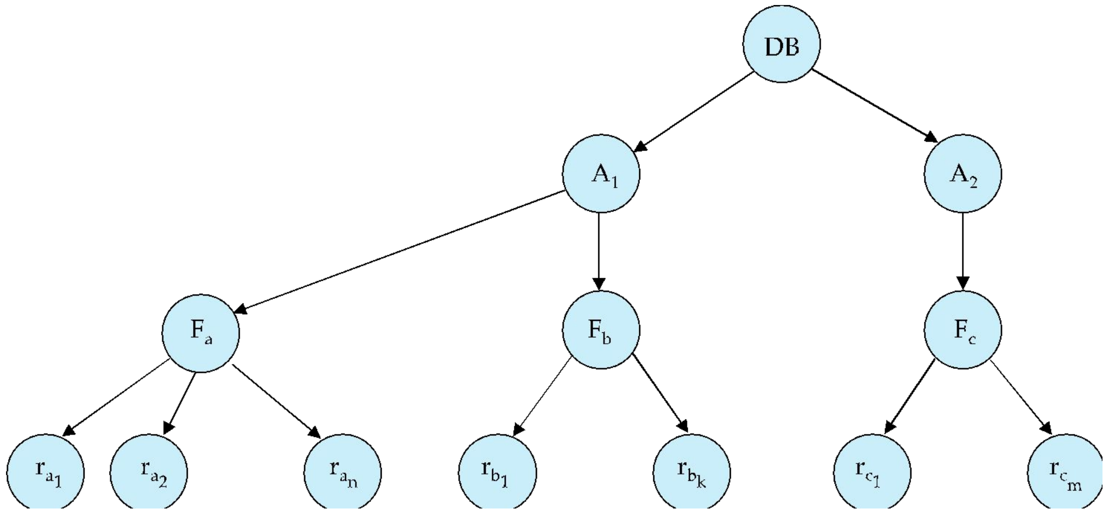
- **Timeout-Based Schemes**:
  - A transaction waits for a lock only for a specified amount of time. After that, the wait times out and the transaction is rolled back.
  - Ensures that deadlocks get resolved by timeout if they occur
  - Simple to implement
  - But may roll back transaction unnecessarily in absence of deadlock
    - Difficult to determine good value of the timeout interval.
  - Starvation is also possible

Multiple Granularity

- Allow data items to be of various sizes and define a hierarchy of data granularities, where the small granularities are nested within larger ones
- Can be represented graphically as a tree
- When a transaction locks a node in the tree *explicitly*, it *implicitly* locks all the node's descendants in the same mode.
- Granularity of locking (level in tree where locking is done):
  - **Fine granularity** (lower in tree): high concurrency, high locking overhead
  - **Coarse granularity** (higher in tree): low locking overhead, low concurrency

Example of Granularity Hierarchy

- The levels, starting from the coarsest (top) level are
  - *database*
  - *area*
  - *file*
  - *record*
- The corresponding tree



Intention Lock Modes

- In addition to S and X lock modes, there are three additional lock modes with multiple granularity:
  - ***intention-shared*** (IS): indicates explicit locking at a lower level of the tree but only with shared locks.
  - ***intention-exclusive*** (IX): indicates explicit locking at a lower level with exclusive or shared locks
  - ***shared and intention-exclusive*** (SIX): the subtree rooted by that node is locked explicitly in shared mode and explicit locking is being done at a lower level with exclusive-mode locks.
- Intention locks allow a higher level node to be locked in S or X mode without having to check all descendent nodes.

Compatibility Matrix with Intention Lock Modes

- The compatibility matrix for all lock modes is:

	IS	IX	S	SIX	X
IS	true	true	true	true	false
IX	true	true	false	false	false
S	true	false	true	false	false
SIX	true	false	false	false	false
X	false	false	false	false	false

Multiple Granularity Locking Scheme

- Transaction T_i can lock a node Q, using the following rules:
  1. The lock compatibility matrix must be observed.
  2. The root of the tree must be locked first, and may be locked in any mode.
  3. A node Q can be locked by T_i in S or IS mode only if the parent of Q is currently locked by T_i in either IX or IS mode.
  4. A node Q can be locked by T_i in X, SIX, or IX mode only if the parent of Q is currently locked by T_i in either IX or SIX mode.
  5. T_i can lock a node only if it has not previously unlocked any node (that is, T_i is two-phase).
  6. T_i can unlock a node Q only if none of the children of Q are currently locked by T_i.
- Observe that locks are acquired in root-to-leaf order, whereas they are released in leaf-to-root order.
- **Lock granularity escalation**: in case there are too many locks at a particular level, switch to higher granularity S or X lock

Insert/Delete Operations and Predicate Reads

- Locking rules for insert/delete operations
  - An exclusive lock must be obtained on an item before it is deleted
  - A transaction that inserts a new tuple into the database I automatically given an X-mode lock on the tuple
- Ensures that
  - reads/writes conflict with deletes
  - Inserted tuple is not accessible by other transactions until the transaction that inserts the tuple commits

Phantom Phenomenon

- Example of **phantom phenomenon.**
  - A transaction T1 that performs **predicate read** (or scan) of a relation
    - **select count**(*)
      from *instructor*
      where *dept_name* = 'Physics'
  - and a transaction T2 that inserts a tuple while T1 is active but after predicate read
    - **insert into** *instructor* **values** ('11111', 'Feynman', 'Physics', 94000)

    (conceptually) conflict in spite of not accessing any tuple in common.
- If only tuple locks are used, non-serializable schedules can result
  - E.g. the scan transaction does not see the new instructor, but may read some other tuple written by the update transaction

Insert/Delete Operations and Predicate Reads

- **Another Example**: T1 and T2 both find maximum instructor ID in parallel, and create new instructors with ID = maximum ID + 1
  - Both instructors get same ID, not possible in serializable schedule
- Schedule

T1	T2
Read(instructor where dept_name='Physics')	
	Insert Instructor in Physics
	Insert Instructor in Comp. Sci.
	Commit
Read(instructor where dept_name='Comp. Sci.')	

Handling Phantoms

- There is a conflict at the data level
  - The transaction performing predicate read or scanning the relation is reading information that indicates what tuples the relation contains
  - The transaction inserting/deleting/updating a tuple updates the same information.
  - The conflict should be detected, e.g. by locking the information.
- One solution:
  - Associate a data item with the relation, to represent the information about what tuples the relation contains.
  - Transactions scanning the relation acquire a shared lock in the data item,
  - Transactions inserting or deleting a tuple acquire an exclusive lock on the data item. (Note: locks on the data item do not conflict with locks on individual tuples.)
- Above protocol provides very low concurrency for insertions/deletions.

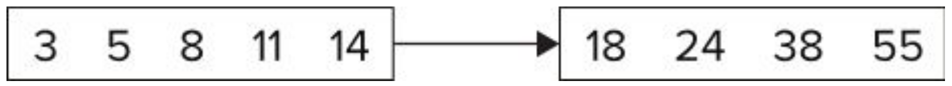
Index Locking To Prevent Phantoms

- **Index locking protocol** to prevent phantoms
  - Every relation must have at least one index.
  - A transaction can access tuples only after finding them through one or more indices on the relation
  - A transaction T_i that performs a lookup must lock all the index leaf nodes that it accesses, in S-mode
    - Even if the leaf node does not contain any tuple satisfying the index lookup (e.g. for a range query, no tuple in a leaf is in the range)
  - A transaction T_i that inserts, updates or deletes a tuple t_i in a relation *r*
    - Must update all indices to *r*
    - Must obtain exclusive locks on all index leaf nodes affected by the insert/update/delete
  - The rules of the two-phase locking protocol must be observed
- Guarantees that phantom phenomenon won't occur

Next-Key Locking to Prevent Phantoms

- Index-locking protocol to prevent phantoms locks entire leaf node
  - Can result in poor concurrency if there are many inserts
- **Next-key locking protocol**: provides higher concurrency
  - Lock all values that satisfy index lookup (match lookup value, or fall in lookup range)
  - Also lock next key value in index
    - even for inserts/deletes
  - Lock mode: S for lookups, X for insert/delete/update
- Ensures detection of query conflicts with inserts, deletes and updates

Consider B+-tree leaf nodes as below, with query predicate $7 \leq X \leq 16$. Check what happens with next-key locking when inserting: (i) 15 and (ii) 7



Timestamp Based Concurrency Control

Timestamp-Based Protocols

- Each transaction T_i is issued a timestamp TS(T_i) when it enters the system.
  - Each transaction has a *unique* timestamp
  - Newer transactions have timestamps strictly greater than earlier ones
  - Timestamp could be based on a logical counter
    - Real time may not be unique
    - Can use (wall-clock time, logical counter) to ensure
- Timestamp-based protocols manage concurrent execution such that
  time-stamp order = serializability order
- Several alternative protocols based on timestamps

Timestamp-Ordering Protocol

The **timestamp ordering (TSO) protocol**

- Maintains for each data *Q* two timestamp values:
  - **W-timestamp**(*Q*) is the largest time-stamp of any transaction that executed **write**(*Q*) successfully.
  - **R-timestamp**(*Q*) is the largest time-stamp of any transaction that executed **read**(*Q*) successfully.
- Imposes rules on read and write operations to ensure that
  - Any conflicting operations are executed in timestamp order
  - Out of order operations cause transaction rollback

Timestamp-Based Protocols (Cont.)

- Suppose a transaction T_i issues a **read**(*Q*)

  1. If TS(T_i) ≤ **W**-timestamp(*Q*), then T_i needs to read a value of *Q* that was already overwritten.
     - Hence, the **read** operation is rejected, and T_i is rolled back.
  2. If TS(T_i) ≥ **W**-timestamp(*Q*), then the **read** operation is executed, and R-timestamp(*Q*) is set to

     max(R-timestamp(*Q*), TS(T_i)).

Timestamp-Based Protocols (Cont.)

- Suppose that transaction T_i issues **write**(Q).
  1. If TS(T_i) < R-timestamp(Q), then the value of Q that T_i is producing was needed previously, and the system assumed that that value would never be produced.
     - Hence, the **write** operation is rejected, and T_i is rolled back.
  2. If TS(T_i) < W-timestamp(Q), then T_i is attempting to write an obsolete value of Q.
     - Hence, this **write** operation is rejected, and T_i is rolled back.
  3. Otherwise, the **write** operation is executed, and W-timestamp(Q) is set to TS(T_i).

Example of Schedule Under TSO

- Is this schedule valid under TSO?

  Assume that initially:
  R-TS(A) = W-TS(A) = 0
  R-TS(B) = W-TS(B) = 0
  Assume TS(T_{25}) = 25 and
  TS(T_{26}) = 26

T_{25}	T_{26}
read(B)	
	read(B)
	$B := B - 50$
	write(B)
read(A)	
	read(A)
display($A + B$)	
	$A := A + 50$
	write(A)
	display($A + B$)

- How about this one, where initially
  R-TS(Q)=W-TS(Q)=0

T_{27}	T_{28}
read(Q)	
	write(Q)
write(Q)	

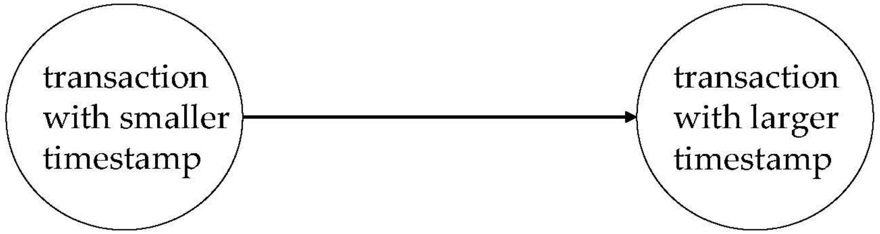
Another Example Under TSO

A partial schedule for several data items for transactions with timestamps 1, 2, 3, 4, 5, with all R-TS and W-TS = 0 initially

T_1	T_2	T_3	T_4	T_5
				read (X)
	read (Y)			
read (Y)				
		write (Y)		
		write (Z)		
				read (Z)
	read (Z)			
	abort			
read (X)				
			read (W)	
		write (W)		
		abort		
				write (Y)
				write (Z)

Correctness of Timestamp-Ordering Protocol

- The timestamp-ordering protocol guarantees serializability since all the arcs in the precedence graph are of the form:


transaction with smaller timestamp
transaction with larger timestamp

Thus, there will be no cycles in the precedence graph

- Timestamp protocol ensures freedom from deadlock as no transaction ever waits.
- But the schedule may not be cascade-free, and may not even be recoverable.

Recoverability and Cascade Freedom

- Solution 1:
  - A transaction is structured such that its writes are all performed at the end of its processing
  - All writes of a transaction form an atomic action; no transaction may execute while a transaction is being written
  - A transaction that aborts is restarted with a new timestamp
- Solution 2:
  - Limited form of locking: wait for data to be committed before reading it
- Solution 3:
  - Use commit dependencies to ensure recoverability

Thomas' Write Rule

- Modified version of the timestamp-ordering protocol in which obsolete **write** operations may be ignored under certain circumstances.
- When T_i attempts to write data item Q, if TS(T_i) < W-timestamp(Q), then T_i is attempting to write an obsolete value of {Q}.
  - Rather than rolling back T_i as the timestamp ordering protocol would have done, this {**write**} operation can be ignored.
- Otherwise this protocol is the same as the timestamp ordering protocol.
- Thomas' Write Rule allows greater potential concurrency.
  - Allows some view-serializable schedules that are not conflict-serializable.

Validation-Based Protocol

- Idea: can we use commit time as serialization order?
- To do so:
  - Postpone writes to end of transaction
  - Keep track of data items read/written by transaction
  - **Validation** performed at commit time, detect any out-of-serialization order reads/writes
- Also called as **optimistic concurrency control** since transaction executes fully in the hope that all will go well during validation

Validation-Based Protocol

- Execution of transaction T_i is done in three phases.

1. **Read and execution phase**: Transaction T_i writes only to temporary local variables
2. **Validation phase**: Transaction T_i performs a ''‘validation test'' to determine if local variables can be written without violating serializability.
3. **Write phase**: If T_i is validated, the updates are applied to the database; otherwise, T_i is rolled back.

- The three phases of concurrently executing transactions can be interleaved, but each transaction must go through the three phases in that order.
  - We assume for simplicity that the validation and write phase occur together, atomically and serially
    - I.e., only one transaction executes validation/write at a time.

Validation-Based Protocol (Cont.)

- Each transaction T_i has 3 timestamps
  - **StartTS**(T_i) : the time when T_i started its execution
  - **ValidationTS**(T_i): the time when T_i entered its validation phase
  - **FinishTS**(T_i) : the time when T_i finished its write phase
- Validation tests use above timestamps and read/write sets to ensure that serializability order is determined by validation time
  - Thus, TS(T_i) = ValidationTS(T_i)
- Validation-based protocol has been found to give greater degree of concurrency than locking/TSO if probability of conflicts is low.

Validation Test for Transaction T_j

- If for all T_i with TS (T_i) < TS (T_j) either one of the following condition holds:
  - **finishTS**(T_i) < **startTS**(T_j)
  - **startTS**(T_j) < **finishTS**(T_i) < **validationTS**(T_j) **and** the set of data items written by T_i does not intersect with the set of data items read by T_j.

  then validation succeeds and T_j can be committed.
- Otherwise, validation fails and T_j is aborted.
- Justification:
  - First condition applies when execution is not concurrent
    - The writes of T_j do not affect reads of T_i since they occur after T_i has finished its reads.
  - If the second condition holds, execution is concurrent, T_j does not read any item written by T_i.

Schedule Produced by Validation

- Example of schedule produced using validation

T_{25}	T_{26}
read(B)	
	read(B)
	$B := B - 50$
	read(A)
	$A := A + 50$
read(A)	
<validate>	
display($A + B$)	
	<validate>
	write(B)
	write(A)

Multiversion Concurrency Control

Multiversion Schemes

- Multiversion schemes keep old versions of data item to increase concurrency. Several variants:
  - **Multiversion Timestamp Ordering**
  - **Multiversion Two-Phase Locking**
  - **Snapshot isolation**
- Key ideas:
  - Each successful **write** results in the creation of a new version of the data item written.
  - Use timestamps to label versions.
  - When a **read**(*Q*) operation is issued, select an appropriate version of *Q* based on the timestamp of the transaction issuing the read request, and return the value of the selected version.
- **read**s never have to wait as an appropriate version is returned immediately.

Multiversion Timestamp Ordering

- Each data item *Q* has a sequence of versions <Q_1, Q_2,...., Q_m>. Each version Q_k contains three data fields:
  - **Content** -- the value of version Q_k.
  - **W-timestamp**(Q_k) -- timestamp of the transaction that created (wrote) version Q_k
  - **R-timestamp**(Q_k) -- largest timestamp of a transaction that successfully read version Q_k

Multiversion Timestamp Ordering (Cont)

- Suppose that transaction T_i issues a **read**(Q) or **write**(Q) operation. Let Q_k denote the version of Q whose write timestamp is the largest write timestamp less than or equal to TS(T_i).
  1. If transaction T_i issues a **read**(Q), then
     - the value returned is the content of version Q_k
     - If R-timestamp(Q_k) < TS(T_i), set R-timestamp(Q_k) = TS(T_i),
  2. If transaction T_i issues a **write**(Q)
     1. if TS(T_i) < R-timestamp(Q_k), then transaction T_i is rolled back.
     2. if TS(T_i) = W-timestamp(Q_k), the contents of Q_k are overwritten
     3. Otherwise, a new version Q_i of Q is created
        - W-timestamp(Q_i) and R-timestamp(Q_i) are initialized to TS(T_i).

Multiversion Timestamp Ordering (Cont)

- Observations
  - Reads always succeed
  - A write by T_i is rejected if some other transaction T_j that (in the serialization order defined by the timestamp values) should read T_i's write, has already read a version created by a transaction older than T_i.
- Protocol guarantees serializability

Multiversion Two-Phase Locking

- Differentiates between read-only transactions and update transactions
- **Update transactions** acquire read and write locks, and hold all locks up to the end of the transaction. That is, update transactions follow rigorous two-phase locking.
  - Read of a data item returns the latest version of the item
  - The first **write** of Q by T_i results in the creation of a new version Q_i of the data item Q written
    - W-timestamp(Q_i) set to ∞ initially
  - When update transaction T_i completes, commit processing occurs:
    - Value **ts-counter** stored in the database is used to assign timestamps
      - **ts-counter** is locked in two-phase manner
    - Set TS(T_i) = **ts-counter** + 1
    - Set W-timestamp(Q_i) = TS(T_i) for all versions Q_i that it creates
    - **ts-counter** = **ts-counter + 1**

Multiversion Two-Phase Locking (Cont.)

- **Read-only transactions**
  - are assigned a timestamp = **ts-counter** when they start execution
  - follow the multiversion timestamp-ordering protocol for performing reads
    - Do not obtain any locks
- Read-only transactions that start after T_i increments **ts-counter** will see the values updated by T_i.
- Read-only transactions that start before T_i increments the **ts-counter** will see the value before the updates by T_i.
- Only serializable schedules are produced.

MVCC: Implementation Issues

- Creation of multiple versions increases storage overhead
  - Extra tuples
  - Extra space in each tuple for storing version information
- Versions can, however, be garbage collected
  - E.g., if Q has two versions Q5 and Q9, and the oldest active transaction has timestamp > 9, than Q5 will never be required again
- Issues with
  - primary key and foreign key constraint checking
  - Indexing of records with multiple versions

  See textbook for details

Snapshot Isolation

- Motivation: Decision support queries that read large amounts of data have concurrency conflicts with OLTP transactions that update a few rows
  - Poor performance results
- Solution 1: Use multiversion 2-phase locking
  - Give logical "snapshot" of database state to read only transaction
    - Reads performed on snapshot
  - Update (read-write) transactions use normal locking
  - Works well, but how does system know a transaction is read only?
- Solution 2 (partial): Give snapshot of database state to every transaction
  - Reads performed on snapshot
  - Use 2-phase locking on updated data items
  - Problem: variety of anomalies such as lost update can result
  - Better solution: snapshot isolation level (next slide)

Snapshot Isolation

- A transaction T1 executing with Snapshot Isolation
  - Takes snapshot of committed data at start
  - Always reads/modifies data in its own snapshot
  - Updates of concurrent transactions are not visible to T1
  - Writes of T1 complete when it commits
  - **First-committer-wins rule**:
    4 Commits only if no other concurrent transaction has already written data that T1 intends to write.

T1	T2	T3
W(Y := 1) Commit		
	Start R(X) □ 0 R(Y)□ 1	

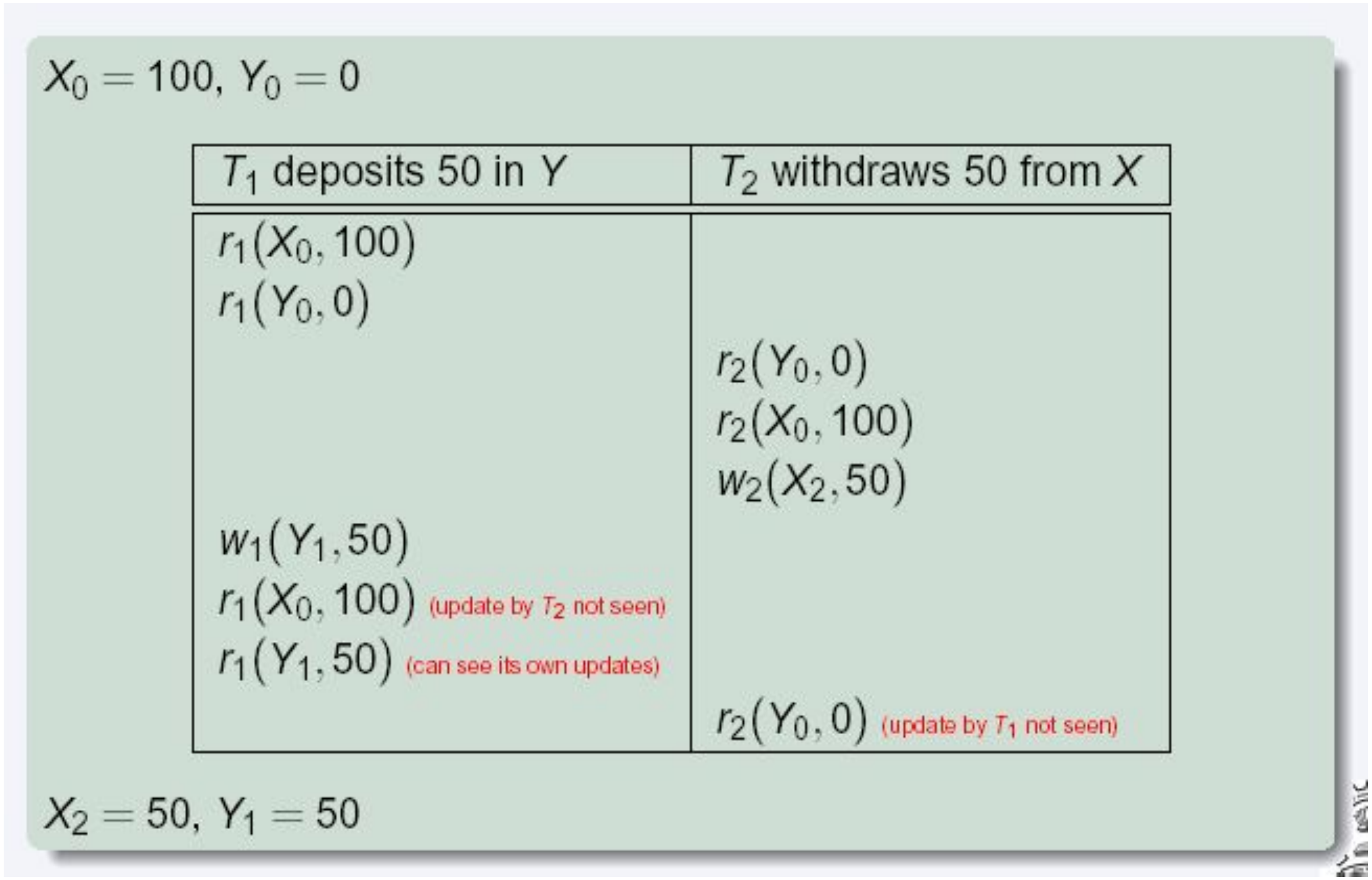
		W(X:=2)
		W(Z:=3)
		Commit

	R(Z) □ 0 R(Y) □ 1 W(X:=3) Commit-Req Abort	

Concurrent updates not visible
Own updates are visible
Not first-committer of X
Serialization error, T2 is rolled back

Snapshot Read

- Concurrent updates invisible to snapshot read

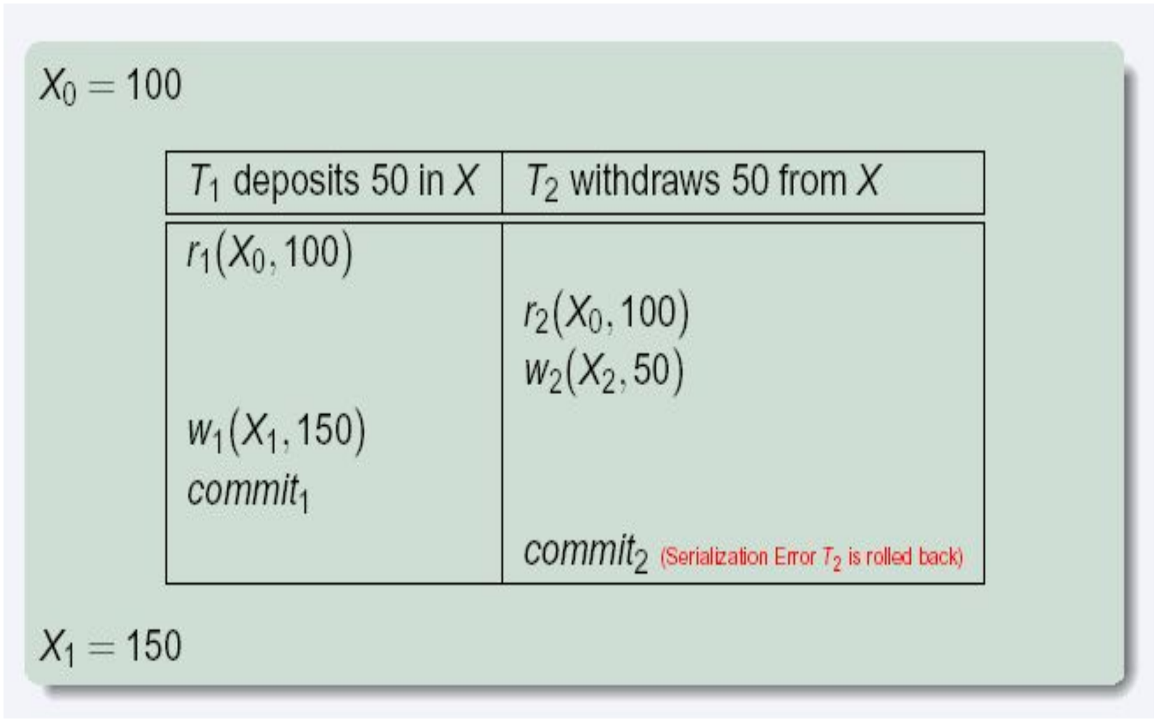


$X_0 = 100,\ Y_0 = 0$

T_1 deposits 50 in Y	T_2 withdraws 50 from X
$r_1(X_0, 100)$	
$r_1(Y_0, 0)$	
	$r_2(Y_0, 0)$
	$r_2(X_0, 100)$
	$w_2(X_2, 50)$
$w_1(Y_1, 50)$	
$r_1(X_0, 100)$ (update by T_2 not seen)	
$r_1(Y_1, 50)$ (can see its own updates)	
	$r_2(Y_0, 0)$ (update by T_1 not seen)

$X_2 = 50,\ Y_1 = 50$

Snapshot Write: First Committer Wins



$X_0 = 100$

T_1 deposits 50 in X	T_2 withdraws 50 from X
$r_1(X_0, 100)$	
	$r_2(X_0, 100)$
	$w_2(X_2, 50)$
$w_1(X_1, 150)$	
$commit_1$	
	$commit_2$ (Serialization Error T_2 is rolled back)

$X_1 = 150$

- Variant: “**First-updater-wins**”
  - Check for concurrent updates when write occurs by locking item
    - 4 But lock should be held till all concurrent transactions have finished
  - (Oracle uses this plus some extra features)
  - Differs only in when abort occurs, otherwise equivalent

Benefits of SI

- Reads are *never* blocked,
  - and also don't block other txns activities
- Performance similar to Read Committed
- Avoids several anomalies
  - No dirty read, i.e. no read of uncommitted data
  - No lost update
    - I.e., update made by a transaction is overwritten by another transaction that did not see the update)
  - No non-repeatable read
    - I.e., if read is executed again, it will see the same value
- Problems with SI
  - SI does not always give serializable executions
    - Serializable: among two concurrent txns, one sees the effects of the other
    - In SI: neither sees the effects of the other
  - Result: Integrity constraints can be violated

Snapshot Isolation

- Example of problem with SI
  - Initially A = 3 and B = 17
    - Serial execution: A = ??, B = ??
    - if both transactions start at the same time, with snapshot isolation: A = ?? , B = ??
- Called **skew write**
- Skew also occurs with inserts
  - E.g:
    - Find max order number among all orders
    - Create a new order with order number = previous max + 1
    - Two transaction can both create order with same number
      - Is an example of phantom phenomenon

T_i	T_j
read(A)	
read(B)	
	read(A)
	read(B)
A=B	
	B=A
write(A)	
	write(B)

Snapshot Isolation Anomalies

- SI breaks serializability when transactions modify *different* items, each based on a previous state of the item the other modified
  - Not very common in practice
    - E.g., the TPC-C benchmark runs correctly under SI
    - when txns conflict due to modifying different data, there is usually also a shared item they both modify, so SI will abort one of them
  - But problems do occur
    - Application developers should be careful about write skew
- SI can also cause a read-only transaction anomaly, where read-only transaction may see an inconsistent state even if updaters are serializable
  - We omit details
- Using snapshots to verify primary/foreign key integrity can lead to inconsistency
  - Integrity constraint checking usually done outside of snapshot

Serializable Snapshot Isolation

- **Serializable snapshot isolation (SSI)**: extension of snapshot isolation that ensures serializability
- Snapshot isolation tracks write-write conflicts, but does not track read-write conflicts
  - Where T_i writes a data a data item Q, T_j reads an earlier version of Q, but T_j is serialized after T_i
- Idea: track read-write dependencies separately, and roll-back transactions where cycles can occur
  - Ensures serializability
  - Details in book
- Implemented in PostgreSQL from version 9.1 onwards
  - PostgreSQL implementation of SSI also uses index locking to detect phantom conflicts, thus ensuring true serializability

SI Implementations

- Snapshot isolation supported by many databases
  - Including Oracle, PostgreSQL, SQL Server, IBM DB2, etc
  - Isolation level can be set to snapshot isolation
- Oracle implements “first updater wins” rule (variant of “first committer wins”)
  - Concurrent writer check is done at time of write, not at commit time
  - Allows transactions to be rolled back earlier
- **Warning**: *even if isolation level is set to serializable, Oracle actually uses snapshot isolation*
  - Old versions of PostgreSQL prior to 9.1 did this too
  - Oracle and PostgreSQL < 9.1 do not support true serializable execution

Working Around SI Anomalies

- Can work around SI anomalies for specific queries by using **select .. for update** (supported e.g. in Oracle)
  - Example
    - **select max**(orderno) **from** orders **for update**
    - read value into local variable maxorder
    - insert into orders (maxorder+1, …)
- **select for update (SFU) clause** treats all data read by the query as if it were also updated, preventing concurrent updates
- Can be added to queries to ensure serializability in many applications
  - Does not handle phantom phenomenon/predicate reads though

Weak Levels of Concurrency

Weak Levels of Consistency

- **Degree-two consistency**: differs from two-phase locking in that S-locks may be released at any time, and locks may be acquired at any time
  - X-locks must be held till end of transaction
  - Serializability is not guaranteed, programmer must ensure that no erroneous database state will occur]
- **Cursor stability**:
  - For reads, each tuple is locked, read, and lock is immediately released
  - X-locks are held till end of transaction
  - Special case of degree-two consistency

Weak Levels of Consistency in SQL

- SQL allows non-serializable executions
  - **Serializable**: is the default
  - **Repeatable read**: allows only committed records to be read, and repeating a read should return the same value (so read locks should be retained)
    - However, the phantom phenomenon need not be prevented
      - T1 may see some records inserted by T2, but may not see others inserted by T2
  - **Read committed**: same as degree two consistency, but most systems implement it as cursor-stability
  - **Read uncommitted**: allows even uncommitted data to be read
- In most database systems, read committed is the default consistency level
  - Can be changed as database configuration parameter, or per transaction
    - **set isolation level serializable**

Concurrency Control across User Interactions

- Many applications need transaction support across user interactions
  - Can’t use locking for long durations
- Application level concurrency control
  - Each tuple has a version number
  - Transaction notes version number when reading tuple
    - **select** r.balance, r.version **into** :A, :version
      from r **where** acctId =23
  - When writing tuple, check that current version number is same as the version when tuple was read
    - **update** r **set** r.balance = r.balance + :deposit, r.version = r.version+1
      where acctId = 23 **and** r.version = :version

Concurrency Control across User Interactions

- Equivalent to **optimistic concurrency control without validating read set**
  - Unlike SI, reads are not guaranteed to be from a single snapshot.
  - Does not guarantee serializability
  - But avoids some anomalies such as “lost update anomaly”
- Used internally in Hibernate ORM system
- Implemented manually in many applications
- Version numbers stored in tuples can also be used to support first committer wins check of snapshot isolation

Advanced topics in Concurrency Control

Online Index Creation

- Problem: how to create an index on a large relation without affecting concurrent updates
  - Index construction may take a long time
  - Two-phase locking will block all concurrent updates
- Key ideas:
  - Build index on a snapshot of the relation, but keep track of all updates that occur after snapshot
    - Updates are not applied on the index at this point
  - Then apply subsequent updates to catch up
  - Acquire relation lock towards end of catchup phase to block concurrent updates
  - Catch up with remaining updates, and add index to system catalog
  - Subsequent transactions will find the index in catalog and update it

Concurrency in Index Structures

- Indices are unlike other database items in that their only job is to help in accessing data.
- Index-structures are typically accessed very often, much more than other database items.
  - Treating index-structures like other database items, e.g. by 2-phase locking of index nodes can lead to low concurrency.
- There are several index concurrency protocols where locks on internal nodes are released early, and not in a two-phase fashion.
  - It is acceptable to have nonserializable concurrent access to an index as long as the accuracy of the index is maintained.
    - In particular, the exact values read in an internal node of a B^+-tree are irrelevant so long as we land up in the correct leaf node.

Concurrency in Index Structures (Cont.)

- **Crabbing protocol** used instead of two-phase locking on the nodes of the B^+-tree during search/insertion/deletion:
  - First lock the root node in shared mode.
  - After locking all required children of a node in shared mode, release the lock on the node
  - During insertion/deletion, upgrade leaf node locks to exclusive mode.
  - When splitting or coalescing requires changes to a parent, lock the parent in exclusive mode.
- Above protocol can cause excessive deadlocks
  - Searches coming down the tree deadlock with updates going up the tree
  - Can abort and restart search, without affecting transaction
- The **B-link tree locking protocol** improves concurrency
  - Intuition: release lock on parent before acquiring lock on child
    - And deal with changes that may have happened between lock release and acquire

Concurrency Control in Main-Memory Databases

- Index locking protocols can be simplified with main-memory databases
  - Short term lock can be obtained on entire index for duration of an operation, serializing updates on the index
    - Avoids overheads of multiple lock acquire/release
    - No major penalty since operations finish fast, since there is no disk wait
- Latch-free techniques for data-structure update can speed up operations further

Latch-Free Data-structure Updates

- This code is not safe without latches if executed concurrently:

```
insert(value, head) {
    node = new node
    node–>value = value
    node–>next = head
    head = node
}
```

- This code is safe

```
insert latchfree(head, value) {
    node = new node
    node–>value = value
    repeat
        oldhead = head
        node–>next = oldhead
        result = CAS(head, oldhead, node)
    until (result == success)
}
```

Latch-Free Data-structure Updates

- This code is not safe without latches if executed concurrently:

```
insert(value, head) {
    node = new node
    node–>value = value
    node–>next = head
    head = node
}
```

- This code is safe

```
insert latchfree(head, value) {
    node = new node
    node–>value = value
    repeat
        oldhead = head
        node–>next = oldhead
        result = CAS(head, oldhead, node)
    until (result == success)
}
```

Latch-Free Data-structures (Cont.)

- *Consider:*

```
delete latchfree(head) {
    /* This function is not quite safe; see explanation in text. */
    repeat
        oldhead = head
        newhead = oldhead–>next
        result = CAS(head, oldhead, newhead)
    until (result == success)
}
```

- Above code is almost correct, but has a concurrency bug
  - P1 initiates delete with N1 as head; concurrently P2 deletes N1 and next node N2, and then reinserts N1 as head, with N3 as next
  - P1 may set head as N2 instead of N3.
- Known as ABA problem
- See book for details of how to avoid this problem

Concurrency Control with Operations

- Consider this non-two phase schedule, which preserves database integrity constraints
- Can be understood as transaction performing increment operation
  - E.g., increment(A, -50), increment (B, 50)
  - As long as increment operation does not return actual value, increments can be reordered
    - ***Increments commute***
  - New increment-mode lock to support reordering
  - Conflict matrix with increment lock mode
    - *Two increment operations do not conflict with each other*

T_1	T_2
read(A)	
$A := A - 50$	
write(A)	
	read(B)
	$B := B - 10$
	write(B)
read(B)	
$B := B + 50$	
write(B)	
	read(A)
	$A := A + 10$
	write(A)

	S	X	I
S	true	false	false
X	false	false	false
I	false	false	true

Concurrency Control with Operations (Cont.)

- Undo of increment(v, n) is performed by increment (v, -n)
- Increment_conditional(v, n):
  - Updates v by adding n to it, as long as final v > 0, fails otherwise
  - Can be used to model, e.g. number of available tickets, *avail_tickets,* for a concert
  - Increment_conditional is NOT commutative
    - E.g., last few tickets for a concert
  - But reordering may still be acceptable

Real-Time Transaction Systems

- Transactions in a system may have deadlines within which they must be completed.
  - Hard deadline: missing deadline is an error
  - Firm deadline: value of transaction is 0 in case deadline is missed
  - Soft deadline: transaction still has some value if done after deadline
- Locking can cause blocking
- Optimistic concurrency control (validation protocol) has been shown to do will in a real-time setting

End of Chapter 18