Transactions; Concurrency; Recovery

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Databases

- Data Models
  - Conceptual representation of the data
- Data Retrieval
  - How to ask questions of the database
  - How to answer those questions
- Data Storage
  - How/where to store data, how to access it
- Data Integrity
  - Manage crashes, concurrency
  - Manage semantic inconsistencies

Transaction Concept

- A **transaction** is a *unit* of program execution that accesses and possibly updates various data items.
- E.g. transaction to transfer $50 from account A to account B:
  1. **read**(*A*)
  2. $A := A - 50$
  3. **write**(*A*)
  4. **read**(*B*)
  5. $B := B + 50$
  6. **write**(*B*)
- Two main issues to deal with:
  - Failures of various kinds, such as hardware failures and system crashes
  - Concurrent execution of multiple transactions

Overview

- *Transaction*: A sequence of database actions enclosed within special tags
- Properties:
  - ***A**tomicity*: Entire transaction or nothing
  - ***C**onsistency*: Transaction, executed completely, takes database from one consistent state to another
  - ***I**solation*: Concurrent transactions *appear* to run in isolation
  - ***D**urability*: Effects of committed transactions are not lost
- Consistency: Transaction programmer needs to guarantee that
  - DBMS can do a few things, e.g., enforce constraints on the data
- Rest: DBMS guarantees

How does..

- .. this relate to *queries* that we discussed ?
  - Queries don't update data, so *durability* and *consistency* not relevant
  - Would want *concurrency*
    - Consider a query computing total balance at the end of the day
  - Would want *isolation*
    - What if somebody makes a *transfer* while we are computing the balance
    - Typically not guaranteed for such long-running queries
- TPC-C vs TPC-H

Assumptions and Goals

- Assumptions:
  - The system can crash at any time
  - Similarly, the power can go out at any point
    - Contents of the main memory won't survive a crash, or power outage
  - BUT… **disks are durable. They might stop, but data is not lost.**
    - For now.
  - Disks only guarantee *atomic* <u>*sector*</u> *writes,* nothing more
  - Transactions are by themselves consistent
- Goals:
  - Guaranteed durability, atomicity
  - As much concurrency as possible, while not compromising isolation and/or consistency
    - Two transactions updating the same account balance… NO
    - Two transactions updating different account balances… YES

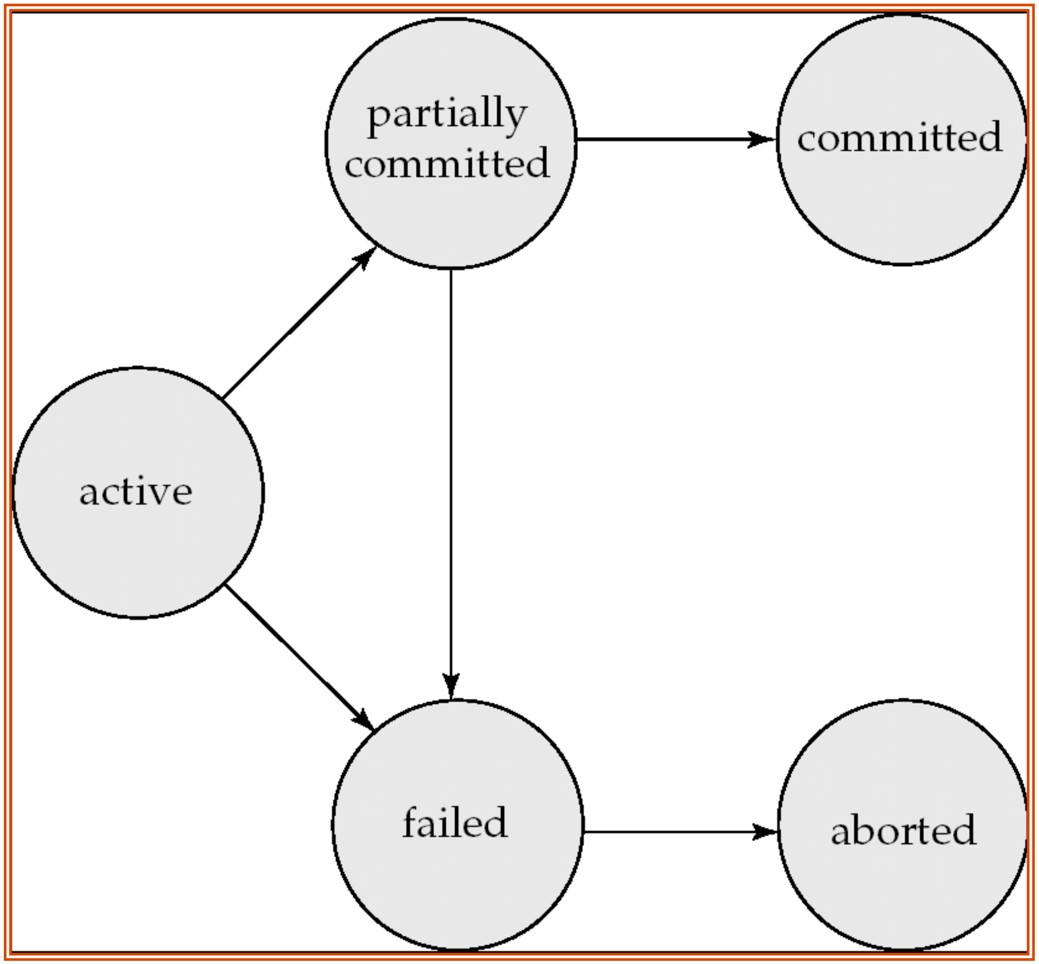
Next…

- States of a transaction

Transaction State

- **Active** – the initial state; the transaction stays in this state while it is executing
- **Partially committed** – after the final statement has been executed.
- **Failed** -- after the discovery that normal execution can no longer proceed.
- **Aborted** – after the transaction has been rolled back and the database restored to its state prior to the start of the transaction. Two options after it has been aborted:
  - restart the transaction
    - can be done only if no internal logical error
  - kill the transaction
- **Committed** – after successful completion.

Transaction states



Next…

- Concurrency: Why?
  - Increased processor and disk utilization
  - Reduced average response times
- Concurrency control schemes
  - A CC scheme is used to guarantee that concurrency does not lead to problems
  - For now, we will assume durability is not a problem
    - So no crashes
    - Though transactions may still abort
- Schedules
- When is concurrency okay ?
  - Serial schedules
  - Serializability

A Schedule

Transactions:

T1: transfers $50 from A to B

T2: transfers 10% of A to B

Database constraint: A + B is constant (*checking+saving accts)*

T1	T2
read(A)	
A = A -50	
write(A)	
read(B)	
B=B+50	
write(B)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	read(B)
	B = B+ tmp
	write(B)

Effect:		Before	After
	A	100	45
	B	50	

105

Each transaction obeys the constraint.

This schedule does too.

Schedules

- A *schedule* is simply a (possibly interleaved) execution sequence of transaction instructions

- *Serial Schedule:* A schedule in which transaction appear one after the other
  - ie., No interleaving

- Serial schedules satisfy isolation and consistency
  - Since each transaction by itself does not introduce inconsistency

Example Schedule

- Another “serial” schedule:

T1	T2
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	read(B)
	B = B+ tmp
	write(B)
read(A)	
A = A -50	
write(A)	
read(B)	
B=B+50	
write(B)	

Effect:	Before	After
A	100	40
B	50	110

Consistent ?

Constraint is satisfied.

Since each Xion is consistent, any serial schedule must be consistent

Another schedule

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Is this schedule okay ?

Lets look at the final effect…

Effect:	Before	After
A	100	45
B	50	105

Consistent.
So this schedule is okay too.

Another schedule

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Is this schedule okay ?

Lets look at the final effect…

Effect:	Before	After
A	100	45
B	50	105

Further, the effect same as the serial schedule 1.

Called *serializable*

Example Schedules (Cont.)

A “bad” schedule

T1	T2
read(A)	
A = A -50	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	read(B)
write(A)	
read(B)	
B=B+50	
write(B)	
	B = B+ tmp
	write(B)

Effect:	<u>Before</u>	<u>After</u>
A	100	50
B	50	60

<u>Not consistent</u>

Serializability

- A schedule is called *serializable* if its final effect is the same as that of a *serial schedule*
- Serializability → schedule is fine and does not result in inconsistent database
  - Since serial schedules are fine
- Non-serializable schedules are unlikely to result in consistent databases
- We will ensure serializability
  - Typically relaxed in real high-throughput environments

Serializability

- Not possible to look at all *n!* serial schedules to check if the effect is the same
  - Instead we ensure serializability by allowing or not allowing certain schedules
- Conflict serializability
- View serializability
  - View serializability allows more schedules

Conflict Serializability

- Two read/write instructions “conflict” if
  - They are by different transactions
  - They operate on the same data item
  - At least one is a “write” instruction

- Why do we care ?
  - If two read/write instructions don’t conflict, they can be “swapped” without any change in the final effect
  - However, if they conflict they CAN’T be swapped without changing the final effect

Equivalence by Swapping

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Effect:	Before	After
A	100	45
B	50	105

==

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
read(B)	
	write(A)
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Effect:	Before	After
A	100	45
B	50	105

Equivalence by Swapping

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Effect:	Before	After
A	100	45
B	50	105

! ==

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
	read(B)
write(B)	
	B = B+ tmp
	write(B)

Effect:	Before	After
A	100	45
B	50	55

Conflict Serializability

- Conflict-equivalent schedules:
  - If S can be transformed into S' through a series of swaps, S and S' are called *conflict-equivalent*
  - *conflict-equivalent guarantees same final effect on the database*

- A schedule S is conflict-serializable if it is conflict-equivalent to a serial schedule

Equivalence by Swapping

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Effect:

	Before	After
A	100	45
B	50	105

==

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
read(B)	
B=B+50	
	write(A)
write(B)	
	read(B)
	B = B+ tmp
	write(B)

Effect:

	Before	After
A	100	45
B	50	105

Equivalence by Swapping

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
read(B)	
B=B+50	
write(B)	
	read(B)
	B = B+ tmp
	write(B)

T1	T2
read(A)	
A = A -50	
write(A)	
read(B)	
B=B+50	
write(B)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	read(B)
	B = B+ tmp
	write(B)

Effect:		Before	After
	A	100	45
	B	50	105

==

Effect:		Before	After
	A	100	45
	B	50	105

Example Schedules (Cont.)

A “bad” schedule

T1	T2
read(A)	
A = A -50	
	Y
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	read(B)
X	
write(A)	
read(B)	
B=B+50	
write(B)	
	B = B+ tmp
	write(B)

Can’t move Y below X
read(B) and write(B) conflict

Other options don’t work either

So: Not Conflict Serializable

Serializability

- In essence, following set of instructions is not conflict-serializable:

T_3	T_4
read(Q)	
	write(Q)
write(Q)	

View-Serializability

- Similarly, following not conflict-serializable

T_3	T_4	T_6
read(Q)		
	write(Q)	
write(Q)		
		write(Q)

- BUT, it is serializable
  - Intuitively, this is because the *conflicting write instructions* don't matter
  - The final write is the only one that matters
- View-serializability allows these
  - Read up

Other notions of serializability

T_1	T_5
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)

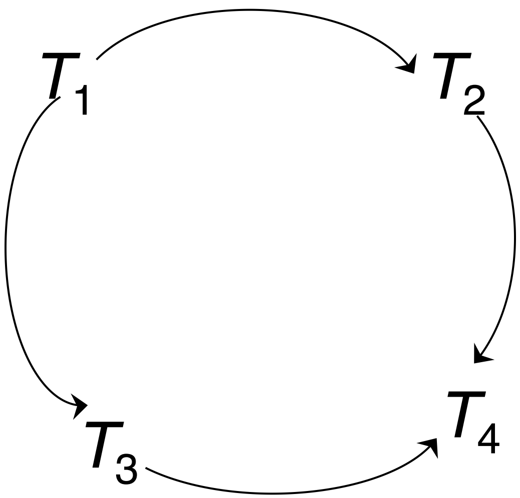
- Not conflict-serializable or view-serializable, but serializable
- Mainly because of the +/- only operations
  - Requires analysis of the actual operations, not just read/write operations
- Most high-performance transaction systems will allow these

Testing for conflict-serializability

- Given a schedule, determine if it is conflict-serializable
- Draw a *precedence-graph* over the transactions
  - A directed edge from T1 and T2, if they have conflicting instructions, and T1's conflicting instruction comes first
- If there is a cycle in the graph, not conflict-serializable
  - Can be checked in at most *O(n+e)* time, where *n* is the number of vertices, and *e* is the number of edges
- If there is none, conflict-serializable
- Testing for view-serializability is NP-hard.

Example Schedule (Schedule A) + Precedence Graph

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



Recap so far…

- We discussed:
  - Serial schedules, serializability
  - Conflict-serializability, view-serializability
  - How to check for conflict-serializability

- We haven't discussed:
  - How to guarantee serializability ?
    - Allowing transactions to run, and then aborting them if the schedules wasn't serializable is clearly not the way to go
  - We instead use schemes to guarantee that the schedule will be conflict-serializable
  - Also, ***recoverability ?***

Recoverability

- Serializability is good for consistency

- But what if transactions fail ?
  - T2 has already committed
    - A user might have been notified
  - Now T1 abort creates a problem
    - T2 has seen its effect, so just aborting T1 is not enough. T2 must be aborted as well (and possibly restarted)
    - But T2 is *committed*

T1	T2
read(A)	
A = A -50	
write(A)	
	read(A)
	tmp = A*0.1
	A = A – tmp
	write(A)
	COMMIT
read(B)	
B=B+50	
write(B)	
ABORT	

Recoverability

- Recoverable schedule: If T1 has read something T2 has written, T2 must commit before T1
  - Otherwise, if T1 commits, and T2 aborts, we have a problem

- Cascading rollbacks: If T10 aborts, T11 must abort, and hence T12 must abort and so on.

T_{10}	T_{11}	T_{12}
read(A)		
read(B)		
write(A)		
	read(A)	
	write(A)	
		read(A)

Recoverability

- *Dirty read*: Reading a value written by a transaction that hasn't committed yet
- Cascadeless schedules:
  - A transaction only reads *committed* values.
  - So if T1 has written A, but not committed it, T2 can't read it.
    - *No dirty reads*
- Cascadeless → No cascading rollbacks
  - That's good
  - We will try to guarantee that as well

Recap so far…

- We discussed:
  - Serial schedules, serializability
  - Conflict-serializability, view-serializability
  - How to check for conflict-serializability
  - Recoverability, cascade-less schedules

- We haven't discussed:
  - How to guarantee serializability ?
    - Allowing transactions to run, and then aborting them if the schedules wasn't serializable is clearly not the way to go
  - We instead use schemes to guarantee that the schedule will be conflict-serializable

Concurrency Control

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Approach, Assumptions etc..

- Approach
  - Guarantee conflict-serializability by allowing certain types of concurrency
    - Lock-based
- Assumptions:
  - Durability is not a problem
    - So no crashes
    - Though transactions may still abort
- Goal:
  - Serializability
  - Minimize the bad effect of aborts (cascade-less schedules only)

Lock-based Protocols

- A transaction *must* get a *lock* before operating on the data
- Two types of locks:
  - *Shared* (S) locks (also called *read locks)*
    - Obtained if we want to only read an item
  - *Exclusive* (X) locks (also called *write locks)*
    - Obtained for updating a data item

Lock instructions

- New instructions
  - lock-S: shared lock request
  - lock-X: exclusive lock request
  - unlock: release previously held lock

Example schedule:

T1	T2
read(B)	read(A)
B $\leftarrow$ B-50	read(B)
write(B)	display(A+B)
read(A)	
A $\leftarrow$ A + 50	
write(A)	

Lock instructions

- New instructions
  - lock-S: shared lock request
  - lock-X: exclusive lock request
  - unlock: release previously held lock

Example schedule:

T1	T2
lock-X(B)	lock-S(A)
read(B)	read(A)
B ←B-50	unlock(A)
write(B)	lock-S(B)
unlock(B)	read(B)
lock-X(A)	unlock(B)
read(A)	display(A+B)
A ←A + 50	
write(A)	
unlock(A)	

Lock-based Protocols

- Lock requests are made to the *concurrency control manager*
  - It decides whether to *grant* a lock request
- T1 asks for a lock on data item A, and T2 currently has a lock on it ?
  - Depends

T2 lock type	T1 lock type	Should allow ?
Shared	Shared	YES
Shared	Exclusive	NO
Exclusive	-	NO

- If *compatible,* grant the lock, otherwise T1 waits in a *queue.*

Lock-based Protocols

- How do we actually use this to guarantee serializability/recoverability ?
  - Not enough just to take locks when you need to read/write something

T1

lock-X(B)
read(B)
B $\leftarrow$B-50
write(B)
unlock(B)

lock-X(A)
read(A)
A $\leftarrow$A + 50
write(A)
unlock(A)

lock-X(A), lock-X(B)
A = A-50
B = B+50
unlock(A), unlock(B)

2-Phase Locking Protocol (2PL)

- Phase 1: Growing phase
  - ★ Transaction may obtain locks
  - ★ But may not release them
- Phase 2: Shrinking phase
  - ★ Transaction may only release locks

- Can be shown that this achieves *conflict-serializability*
  - ★ *lock-point*: the time at which a transaction acquired last lock
  - ★ if *lock-point*(T1) < *lock-point*(T2), there can't be an edge from T2 to T1 in the *precedence graph*

T1

lock-X(B)
read(B)
B ← B-50
write(B)
unlock(B)

lock-X(A)
read(A)
A ← A + 50
write(A)
unlock(A)

2 Phase Locking

- Example: T1 in 2PL

	T1
Growing phase	lock-X(B)
	read(B)
	B ← B - 50
	write(B)
	lock-X(A)
	read(A)
	A ← A - 50
	write(A)
Shrinking phase	unlock(B)
	unlock(A)

2 Phase Locking

- Guarantees *conflict-serializability*, but not cascade-less recoverability

T1	T2	T3
lock-X(A), lock-S(B)		
read(A)		
read(B)		
write(A)		
unlock(A), unlock(B)		
	lock-X(A)	
	read(A)	
	write(A)	
	unlock(A)	
	Commit	lock-S(A)
		read(A)
		Commit
<xction fails>		

2 Phase Locking

- Guarantees *conflict-serializability*, but not cascade-less recoverability

- Guaranteeing just recoverability:
  - If T2 reads a dirty data of T1 (ie, T1 has not committed), then T2 can't commit unless T1 either commits or aborts
  - If T1 commits, T2 can proceed with committing
  - If T1 aborts, T2 must abort
    - So cascades still happen

Strict 2PL

- Release *exclusive* locks only at the very end, just before commit or abort

T1	T2	T3
lock-X(A), lock-S(B)		
read(A)		
read(B)		
write(A)		
unlock(A), unlock(B)		
	lock-X(A)	
	read(A)	
	write(A)	
	unlock(A)	
	Commit	lock-S(A)
		read(A)
		Commit
<xction fails>		

Strict 2PL will not allow that

Works. Guarantees cascade-less and recoverable schedules.

Strict 2PL

- Release *exclusive* locks only at the very end, just before commit or abort
  - Read locks are not important

- Rigorous 2PL: Release both *exclusive and read* locks only at the very end
  - The serializability order === the commit order
  - More intuitive behavior for the users
    - No difference for the system

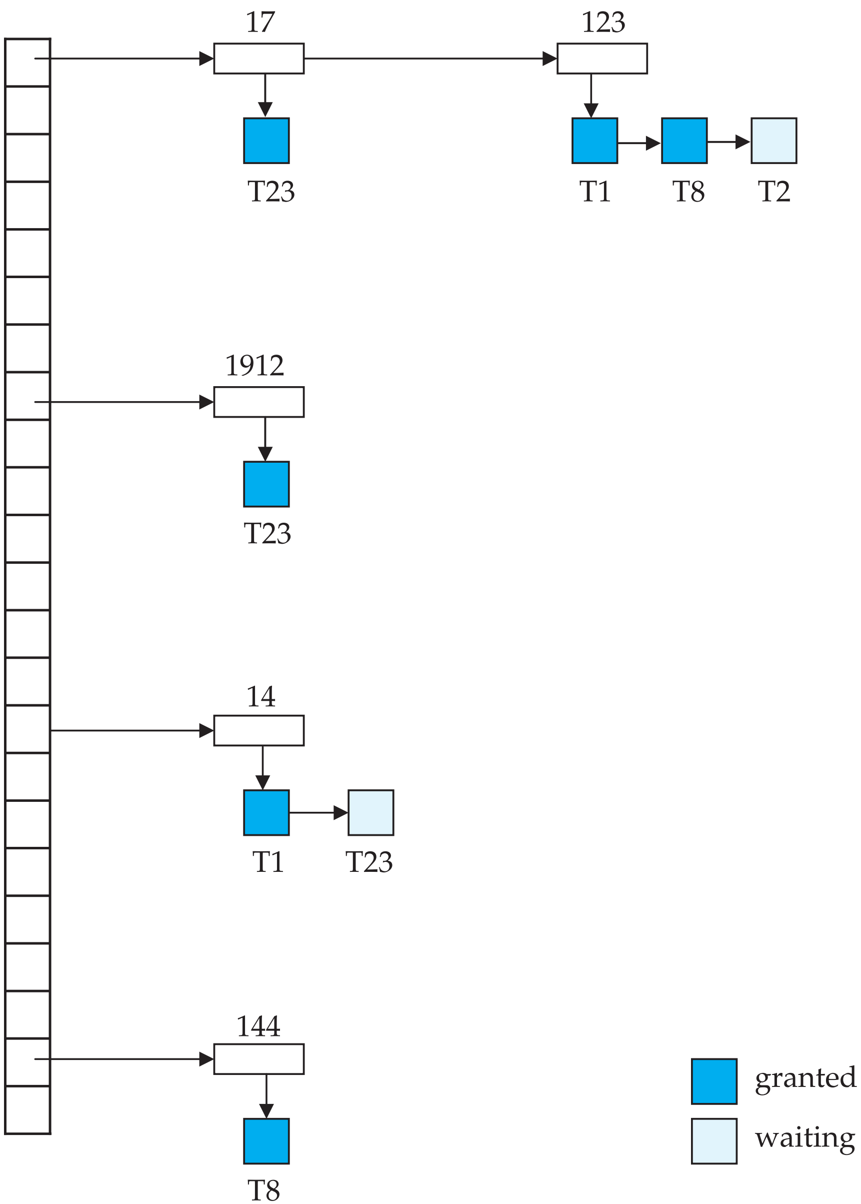
Strict 2PL

- Lock conversion:
  - Transaction might not be sure what it needs a write lock on
  - Start with a S lock
  - *Upgrade* to an X lock later if needed
  - Doesn't change any of the other properties of the protocol

Implementation of Locking

- A separate process, or a separate module
- Uses a *lock table* to keep track of currently assigned locks and the requests for locks
  - Read up in the book

Lock Table



- Black rectangles indicate granted locks, white ones indicate waiting requests
- Lock table also records the type of lock granted or requested
- New request is added to the end of the queue of requests for the data item, and granted if it is compatible with all earlier locks
- Unlock requests result in the request being deleted, and later requests are checked to see if they can now be granted
- If transaction aborts, all waiting or granted requests of the transaction are deleted
  - lock manager may keep a list of locks held by each transaction, to implement this efficiently

Recap so far…

- Concurrency Control Scheme
  - A way to guarantee serializability, recoverability etc
- Lock-based protocols
  - Use *locks* to prevent multiple transactions accessing the same data items
- 2 Phase Locking
  - Locks acquired during *growing phase,* released during *shrinking phase*
- Strict 2PL, Rigorous 2PL

More Locking Issues: Deadlocks

- No xction proceeds:

Deadlock

- T1 waits for T2 to unlock A
- T2 waits for T1 to unlock B

T1	T2
lock-X(B)	
read(B)	
B $\leftarrow$ B-50	
write(B)	
	lock-S(A)
	read(A)
	lock-S(B)
lock-X(A)	

Rollback transactions
Can be costly...

2PL and Deadlocks

- 2PL does not prevent deadlock
  - Strict doesn't either
- > 2 xctions involved?
  - Rollbacks expensive

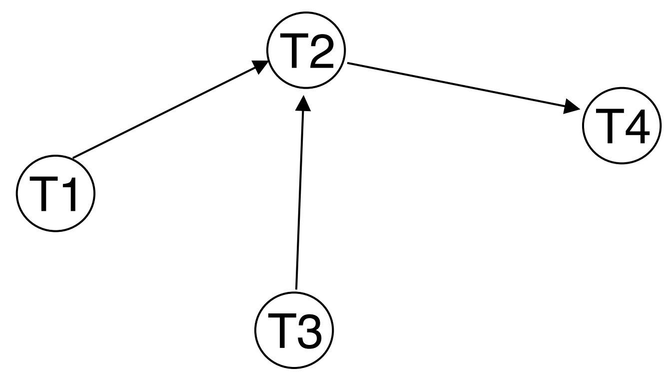
T1	T2
lock-X(B)	
read(B)	
B $\leftarrow$ B-50	
write(B)	
	lock-S(A)
	read(A)
	lock-S(B)
lock-X(A)	

Preventing deadlocks

- Solution 1: A transaction must acquire all locks before it begins
  - Not acceptable in most cases
- Solution 2: A transaction must acquire locks in a particular order over the data items
  - Also called *graph-based protocols*
- Solution 3: Use time-stamps; say T1 is older than T2
  - *wait-die scheme:* T1 will wait for T2. T2 will not wait for T1; instead it will abort and restart
  - *wound-wait scheme:* T1 will *wound* T2 (force it to abort) if it needs a lock that T2 currently has; T2 will wait for T1.
- Solution 4: Timeout based
  - Transaction waits a certain time for a lock; aborts if it doesn't get it by then

Deadlock detection and recovery

- Instead of trying to prevent deadlocks, let them happen and deal with them if they happen
- How do you detect a deadlock?
  - Wait-for graph
  - Directed edge from Ti to Tj
    - Ti waiting for Tj



T1	T2	T3	T4
		X(Z)	
	X(V)		
			X(W)
S(V)			
	S(W)		
		S(V)	

Suppose T4 requests lock-S(Z)....

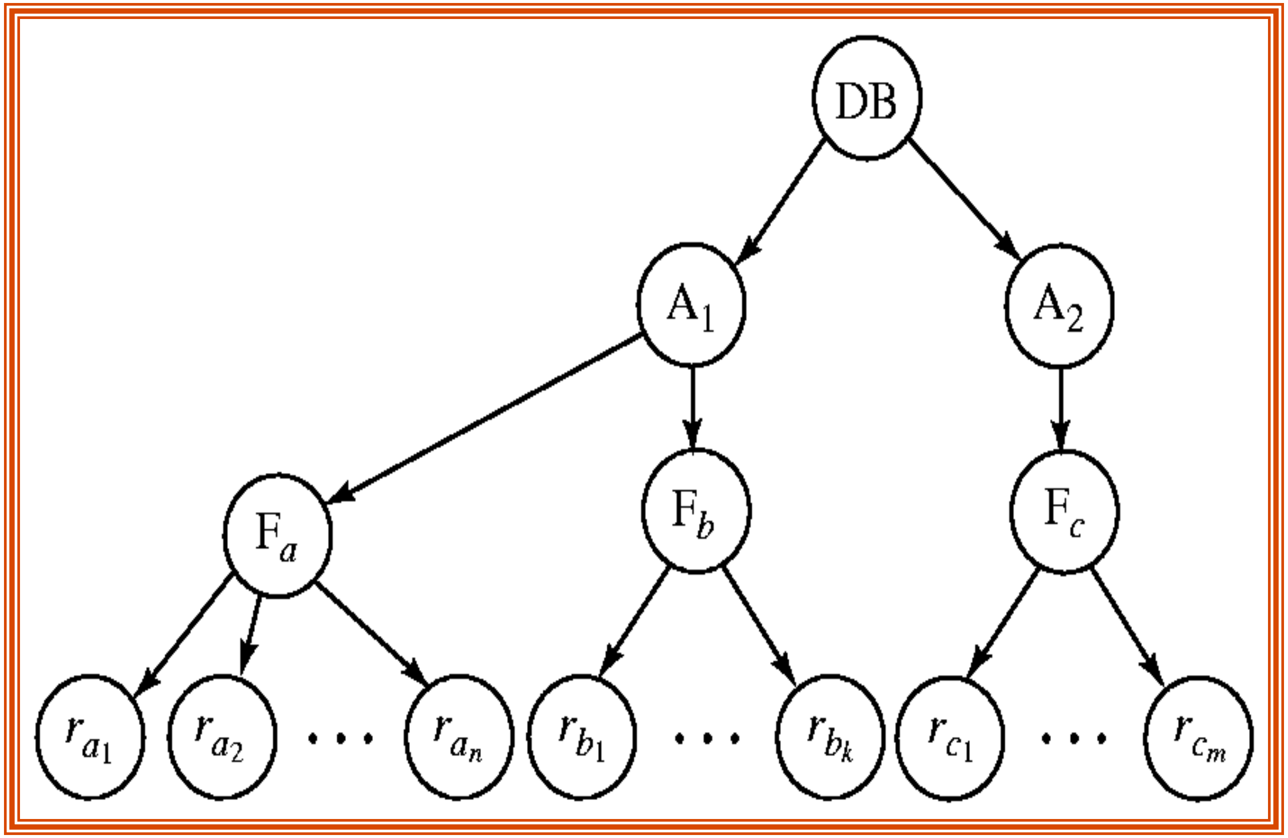
Dealing with Deadlocks

- Deadlock detected, now what ?
  - Will need to abort some transaction
  - Prefer to abort the one with the minimum work done so far
  - Possibility of starvation
    - If a transaction is aborted too many times, it may be given priority in continueing

Locking granularity

- Locking granularity
  - What are we taking locks on ? Tables, tuples, attributes ?
- Coarse granularity
  - e.g. take locks on tables
  - less overhead (the number of tables is not that high)
  - very low concurrency
- Fine granularity
  - e.g. take locks on tuples
  - much higher overhead
  - much higher concurrency
  - What if I want to lock 90% of the tuples of a table ?
    - Prefer to lock the whole table in that case

Granularity Hierarchy



The highest level in the example hierarchy is the entire database.

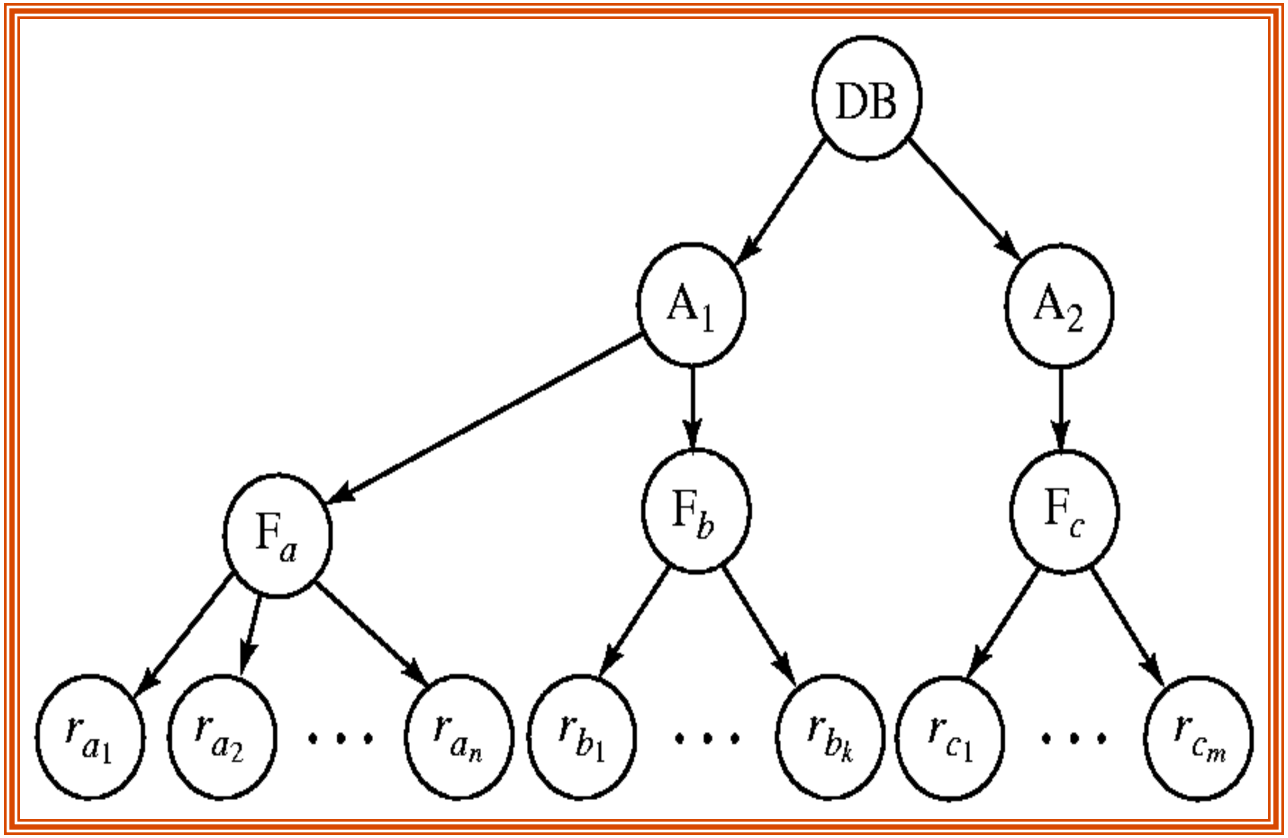
The levels below are of type *area*, *file or relation* and *record* in that order.

Can lock at any level in the hierarchy

Granularity Hierarchy

- New lock mode, called *intentional* locks
  - Declare an intention to lock parts of the subtree below a node
  - IS: *intention shared*
    - The lower levels below may be locked in the shared mode
  - IX: *intention exclusive*
  - *SIX: shared and intention-exclusive*
    - The entire subtree is locked in the shared mode, but I might also want to get exclusive locks on the nodes below
- Protocol:
  - If you want to acquire a lock on a data item, all the ancestors must be locked as well, at least in the intentional mode
  - So you always start at the top *root* node

Granularity Hierarchy

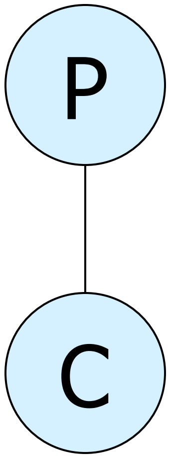


(1) Want to lock *F_a* in shared mode, *DB* and *A1* must be locked in at least IS mode (but IX, SIX, S, X are okay too)

(2) Want to lock *rc1* in exclusive mode, *DB, A2,Fc* must be locked in at least IX mode (SIX, X are okay too)

Granularity Hierarchy

Parent locked in	Child can be locked in
IS	IS, S
IX	IS, S, IX, X, SIX
S	[S, IS] not necessary
SIX	X, IX, [SIX]
X	none



Compatibility Matrix with Intention Lock Modes

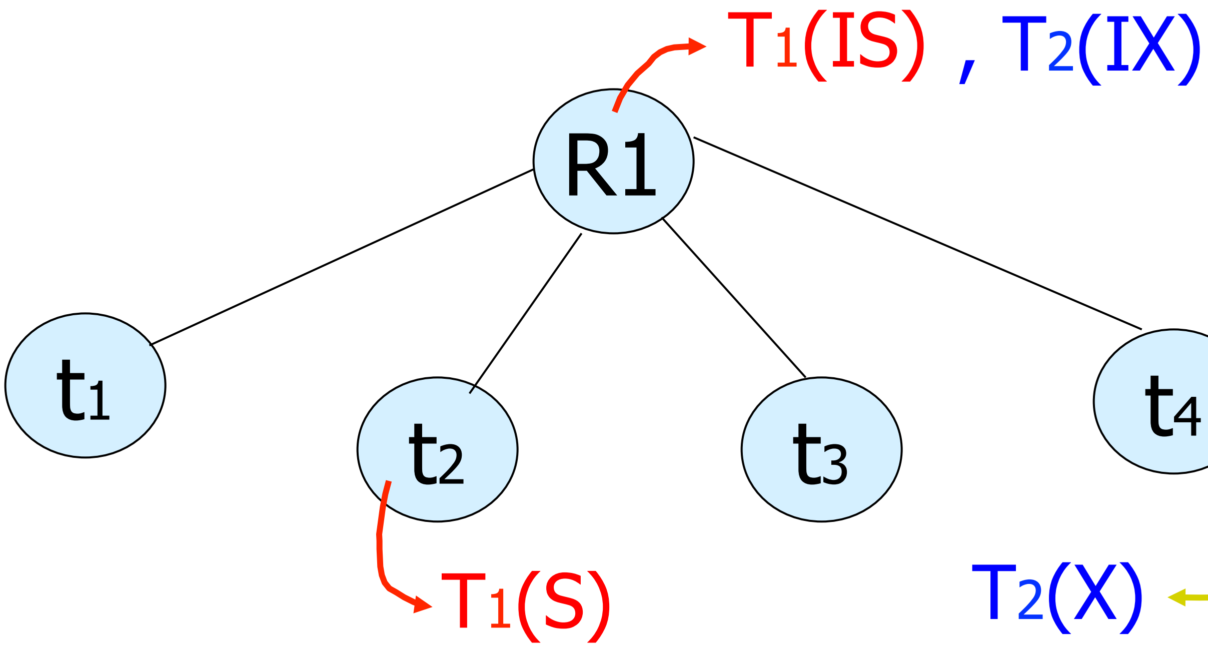
- The compatibility matrix (which locks can be present simultaneously on the same data item) for all lock modes is:

requestor

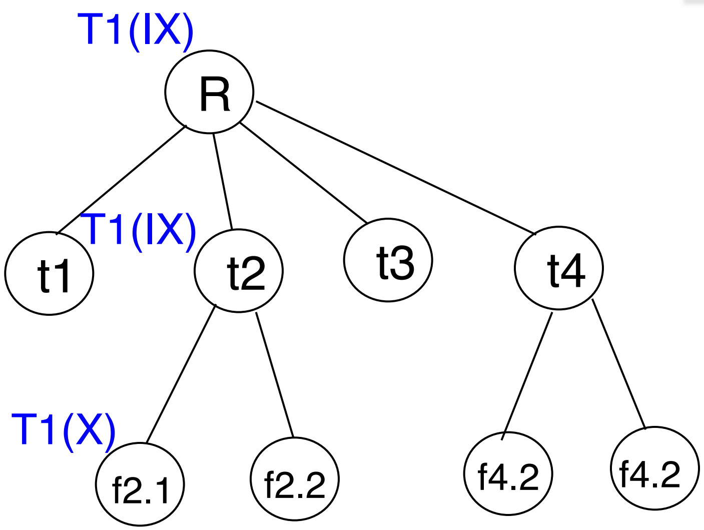
holder

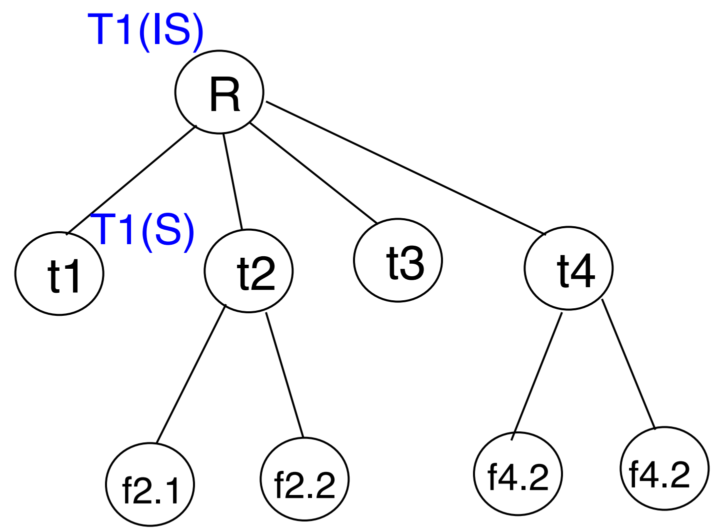
	IS	IX	S	S IX	X
IS	✓	✓	✓	✓	×
IX	✓	✓	×	×	×
S	✓	×	✓	×	×
S IX	✓	×	×	×	×
X	×	×	×	×	×

Example

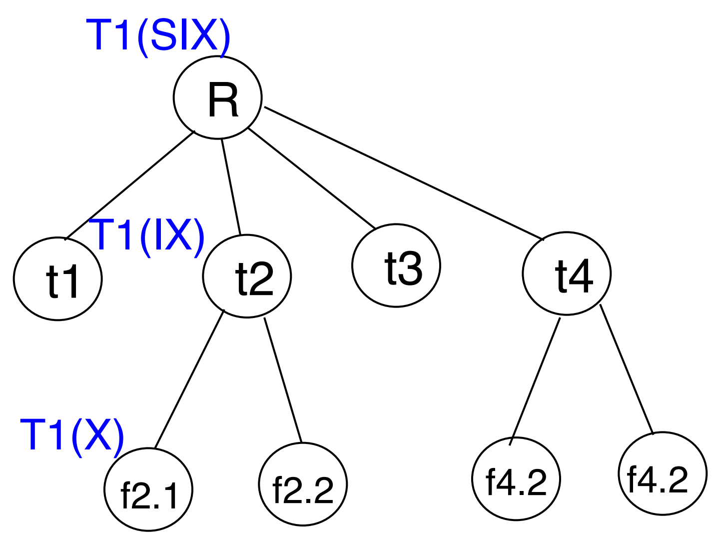


Examples





Can T2 access object f2.2 in X mode?
What locks will T2 get?



Examples

- T1 scans R, and updates a few tuples:
  - T1 gets an SIX lock on R, then repeatedly gets an S lock on tuples of R, and occasionally upgrades to X on the tuples.
- T2 uses an index to read only part of R:
  - T2 gets an IS lock on R, and repeatedly gets an S lock on tuples of R.
- T3 reads all of R:
  - T3 gets an S lock on R.
  - OR, T3 could behave like T2; can

use lock escalation to decide which.

	--	IS	IX	S	X
--	√	√	√	√	√
IS	√	√	√	√	
IX	√	√	√		
S	√	√		√	
X	√				

Recap, Next….

- Deadlocks
  - Detection, prevention, recovery
- Locking granularity
  - Arranged in a hierarchy
  - Intentional locks
- Next…
  - Brief discussion of some other concurrency schemes

Other CC Schemes

- Time-stamp based
  - Transactions are issued time-stamps when they enter the system
  - The time-stamps determine the *serializability* order
  - So if T1 entered before T2, then T1 should be before T2 in the serializability order
  - Say *timestamp(T1) < timestamp(T2)*
  - If T1 wants to read data item A
    - If any transaction with larger time-stamp wrote that data item, then this operation is not permitted, and T1 is *aborted*
  - If T1 wants to write data item A
    - If a transaction with larger time-stamp already read that data item or written it, then the write is *rejected* and T1 is aborted
  - Aborted transaction are restarted with a new timestamp
    - Possibility of *starvation*

Other CC Schemes

- Time-stamp based
  - Example

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

Other CC Schemes

- Time-stamp based
  - As discussed here, has too many problems
    - Starvation
    - Non-recoverable
    - Cascading rollbacks required
  - Most can be solved fairly easily
    - Read up
  - Remember: We can always put more and more restrictions on what the transactions can do to ensure these things
    - The goal is to find the minimal set of restrictions to as to not hinder concurrency

Other Schemes: Optimistic Concurrency Control

- Optimistic concurrency control
  - Also called validation-based
  - Intuition
    - Let the transactions execute as they wish
    - At the very end when they are about to commit, check if there might be any problems/conflicts etc
      - If no, let it commit
      - If yes, abort and restart
  - Optimistic: The hope is that there won't be too many problems/aborts

Other Schemes: Optimistic Concurrency Control

- Each transaction T_i has 3 timestamps
  - Start(T_i) : the time when T_i started its execution
  - Validation(T_i): the time when T_i entered its validation phase
  - Finish(T_i) : the time when T_i finished its write phase
- Serializability order is determined by timestamp given at validation time, to increase concurrency.
  - Thus TS(T_i) is given the value of Validation(T_i).
- This protocol is useful and gives greater degree of concurrency if probability of conflicts is low.
  - because the serializability order is not pre-decided, and
  - relatively few transactions will have to be rolled back.

Other Schemes: Optimistic Concurrency Control

- If for all T_i with TS (T_i) < TS (T_j) either one of the following condition holds:
  - **finish**(T_i) < **start**(T_j)
  - **start**(T_j) < **finish**(T_i) < **validation**(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*: Either the first condition is satisfied, and there is no overlapped execution, or the second condition is satisfied and
  - the writes of T_j do not affect reads of T_i since they occur after T_i has finished its reads.
  - the writes of T_i do not affect reads of T_j since T_j does not read any item written by T_i.

Other Schemes: Optimistic Concurrency Control

- 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)

Other CC Schemes: Snapshot Isolation

- Very popular scheme, used as the primary scheme by many systems including Oracle, PostgreSQL etc…
  - Several others support this in addition to locking-based protocol
- A type of "optimistic concurrency control"
- Key idea:
  - For each object, maintain past "versions" of the data along with timestamps
    - Every update to an object causes a new version to be generated

Other CC Schemes: Snapshot Isolation

- Read queries:
  - Let “t” be the “time-stamp” of the query, i.e., the time at which it entered the system
  - When the query asks for a data item, provide a version of the data item that was latest as of “t”
    - Even if the data changed in between, provide an old version
  - No locks needed, no waiting for any other transactions or queries
  - The query executes on a consistent snapshot of the database
- Update queries (transactions):
  - Reads processed as above on a snapshot
  - Writes are done in private storage
  - At commit time, for each object that was written, check if some other transaction updated the data item since this transaction started
    - If yes, then abort and restart
    - If no, make all the writes public simultaneously (by making new versions)

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**:
    - 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

Other CC Schemes: Snapshot Isolation

- Advantages:
  - Read query don’t block at all, and run very fast
  - As long as conflicts are rare, update transactions don’t abort either
  - Overall better performance than locking-based protocols
- Major disadvantage:
  - Not serializable
  - Inconsistencies may be introduced
  - See the wikipedia article for more details and an example
    - http://en.wikipedia.org/wiki/Snapshot_isolation

Snapshot Isolation

- Example of problem with SI
  - T1: x:=y
  - T2: y:= x
  - Initially x = 3 and y = 17
    - Serial execution: x = ??, y = ??
    - if both transactions start at the same time, with snapshot isolation: x = ?? , y = ??
- 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

SI In Oracle and PostgreSQL

- **Warning**: SI used when isolation level is set to serializable, by Oracle, and PostgreSQL versions prior to 9.1
  - PostgreSQL's implementation of SI (versions prior to 9.1) described in Section 26.4.1.3
  - 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
    - Oracle and PostgreSQL < 9.1 do not support true serializable execution
  - PostgreSQL 9.1 introduced new protocol called "Serializable Snapshot Isolation" (SSI)
    - Which guarantees true serializabilty including handling predicate reads (coming up)

The "Phantom" problem

- An interesting problem that comes up for dynamic databases
- Schema: *accounts(acct_no, balance, zipcode, …)*
- Transaction 1: Find the number of accounts in *zipcode = 20742,* and divide $1,000,000 between them
- Transaction 2: Insert *<acctX, …, 20742, …>*
- Execution sequence:
  - T1 locks all tuples corresponding to "zipcode = 20742", finds the total number of accounts (= num_accounts)
  - T2 does the insert
  - T1 computes 1,000,000/num_accounts
  - When T1 accesses the relation again to update the balances, it finds one new ("phantom") tuples (the new tuple that T2 inserted)
- Not serializable
- See this for another example

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 many database systems, read committed is the default consistency level
  - has to be explicitly changed to serializable when required
    - **set isolation level serializable**

Recovery

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CMSC424

Context

- ACID properties:
  - We have talked about Isolation and Consistency
  - How do we guarantee Atomicity and Durability ?
    - Atomicity: Two problems
      - Part of the transaction is done, but we want to cancel it
        - ABORT/ROLLBACK
      - System crashes during the transaction. Some changes made it to the disk, some didn't.
    - Durability:
      - Transaction finished. User notified. But changes not sent to disk yet (for performance reasons). System crashed.

- Essentially similar solutions

Reasons for crashes

- Transaction failures
  - **Logical errors**: transaction cannot complete due to some internal error condition
  - **System errors**: the database system must terminate an active transaction due to an error condition (e.g., deadlock)
- System crash
  - Power failures, operating system bugs etc
  - **Fail-stop assumption**: non-volatile storage contents are assumed to not be corrupted by system crash
    - Database systems have numerous integrity checks to prevent corruption of disk data
- Disk failure
  - Head crashes; ***for now we will assume***
    - ***STABLE STORAGE: Data never lost. Can approximate by using RAID and maintaining geographically distant copies of the data***

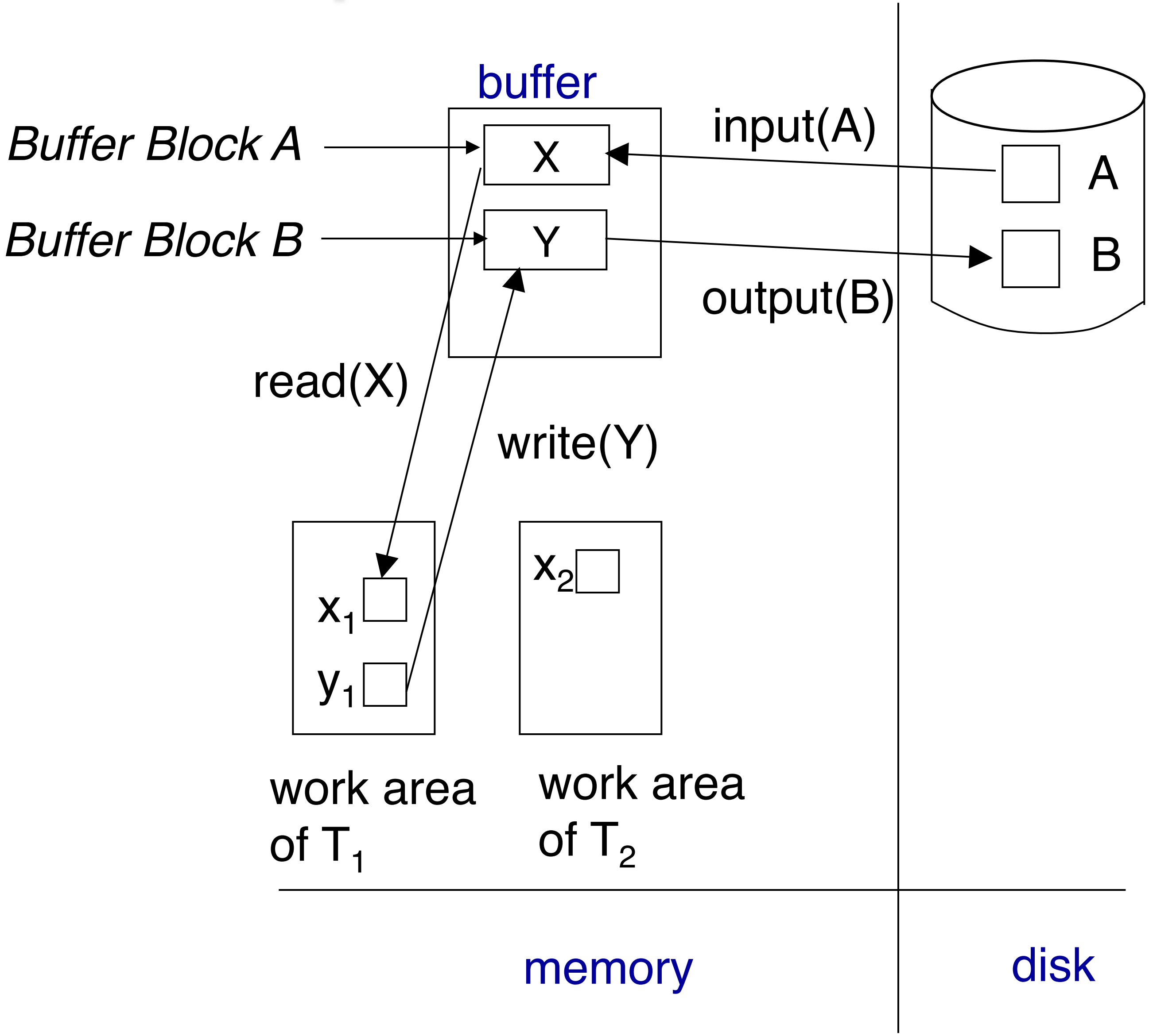
Approach, Assumptions etc..

- Approach:
  - Guarantee A and D:
    - by controlling how the disk and memory interact,
    - by storing enough information during normal processing to recover from failures
    - by developing algorithms to recover the database state
- Assumptions:
  - System may crash, but the *disk is durable*
  - The only *atomicity* guarantee is that *a disk block write* is *atomic*
- Once again, obvious naïve solutions exist that work, but that are too expensive.
  - E.g. The shadow copy solution we saw earlier
    - Make a copy of the database; do the changes on the copy; do an atomic switch of the *dbpointer* at commit time
  - Goal is to do this as efficiently as possible

Data Access

- **Physical blocks** are those blocks residing on the disk.
- **Buffer blocks** are the blocks residing temporarily in main memory.
- Block movements between disk and main memory are initiated through the following two operations:
  - **input**(*B*) transfers the physical block *B* to main memory.
  - **output**(*B*) transfers the buffer block *B* to the disk, and replaces the appropriate physical block there.
- We assume, for simplicity, that each data item fits in, and is stored inside, a single block.

Example of Data Access



Data Access (Cont.)

- Each transaction T_i has its private work-area in which local copies of all data items accessed and updated by it are kept.
  - T_i's local copy of a data item X is called x_i.
- Transferring data items between system buffer blocks and its private work-area done by:
  - **read**(X) assigns the value of data item X to the local variable x_i.
  - **write**(X) assigns the value of local variable x_i to data item $\{X\}$ in the buffer block.
  - **Note: output**(B_X) need not immediately follow **write**(X). System can perform the **output** operation when it deems fit.
- Transactions
  - Must perform **read**(X) before accessing X for the first time (subsequent reads can be from local copy)
  - **write**(X) can be executed at any time before the transaction commits

STEAL vs NO STEAL, FORCE vs NO FORCE

- STEAL:
  - The buffer manager *can steal* a (memory) page from the database
    - ie., it can write an arbitrary page to the disk and use that page for something else from the disk
    - In other words, the database system doesn't control the buffer replacement policy
  - Why a problem ?
    - The page might contain *dirty writes,* ie., writes/updates by a transaction that hasn't committed
  - But, we must allow *steal* for performance reasons.

- NO STEAL:
  - Not allowed. More control, but less flexibility for the buffer manager.

STEAL vs NO STEAL, FORCE vs NO FORCE

- FORCE:
  - The database system *forces* all the updates of a transaction to disk before committing
  - Why ?
    - To make its updates permanent before committing
  - Why a problem ?
    - Most probably random I/Os, so poor response time and throughput
    - Interferes with the disk controlling policies
- NO FORCE:
  - Don't do the above. Desired.
  - Problem:
    - Guaranteeing durability becomes hard
  - We might still have to *force* some pages to disk, but minimal.

STEAL vs NO STEAL, FORCE vs NO FORCE: Recovery implications

	No Steal	Steal
No Force		Desired
Force	Trivial	

STEAL vs NO STEAL, FORCE vs NO FORCE: Recovery implications

- How to implement A and D when No Steal and Force ?
  - Only updates from committed transaction are written to disk (since no steal)
  - Updates from a transaction are forced to disk before commit (since force)
    - A minor problem: how do you guarantee that all updates from a transaction make it to the disk atomically ?
      - Remember we are only guaranteed an atomic *block write*
      - What if some updates make it to disk, and other don't ?
    - Can use something like shadow copying/shadow paging

  - No atomicity/durability problem arise.

Terminology

- Deferred Database Modification:
  - Similar to NO STEAL, NO FORCE
    - Not identical
  - Only need *redos, no undos*
  - We won't cover this today
- Immediate Database Modification:
  - Similar to STEAL, NO FORCE
  - Need both *redos, and undos*

Log-based Recovery

- Most commonly used recovery method
- Intuitively, a log is a record of everything the database system does
- For every operation done by the database, a *log record* is generated and stored *typically on a different (log) disk*
- *<T1, START>*
- *<T2, COMMIT>*
- *<T2, ABORT>*
- *<T1, A, 100, 200>*
  - T1 modified A; old value = 100, new value = 200

Log

- **Example transactions** T_0 **and** T_1 **(**T_0 **executes before** T_1**):**

T_0: **read** (*A*)

A: - A - 50

write (*A*)

read (*B*)

B:- B + 50

write (*B*)

T_1 : **read** (*C*)

C:- C- 100

write (*C*)

- Log:

<T_0 start>	<T_0 start>	<T_0 start>
<T_0, A, 950>	<T_0, A, 950>	<T_0, A, 950>
<T_0, B, 2050>	<T_0, B, 2050>	<T_0, B, 2050>
	<T_0 commit>	<T_0 commit>
	<T_1 start>	<T_1 start>
	<T_1, C, 600>	<T_1, C, 600>
		<T_1 commit>
(a)	(b)	(c)

Log-based Recovery

- Assumptions:
  1. Log records are immediately pushed to the disk as soon as they are generated
  2. Log records are written to disk in the order generated
  3. A log record is generated *before* the actual data value is updated
  4. *Strict two-phase locking*
  - The first assumption can be relaxed
  - As a special case, a transaction is considered *committed* only after the *<T1, COMMIT> has been pushed to the disk*
- But, this seems like exactly what we are trying to avoid ??
  - Log writes are *sequential*
  - They are also typically on a different disk
- Aside: LFS == log-structured file system

Log-based Recovery

- Assumptions:
  1. Log records are immediately pushed to the disk as soon as they are generated
  2. Log records are written to disk in the order generated
  3. A log record is generated _before_ the actual data value is updated
  4. _Strict two-phase locking_
  - ★ The first assumption can be relaxed
  - ★ As a special case, a transaction is considered _committed_ only after the *<T1, COMMIT> has been pushed to the disk*
- NOTE: As a result of assumptions 1 and 2, if *data item A* is updated, the log record corresponding to the update is always forced to the disk before *data item A* is written to the disk
  - ★ This is actually the only property we need; assumption 1 can be relaxed to just guarantee this (called _write-ahead logging)_

Using the log to *abort/rollback*

- STEAL is allowed, so changes of a transaction may have made it to the disk

- UNDO(T1):
  - Procedure executed to *rollback/undo* the effects of a transaction
  - E.g.
    - *<T1, START>*
    - *<T1, A, 200, 300>*
    - *<T1, B, 400, 300>*
    - *<T1, A, 300, 200>* *[[note: second update of A]]*
    - T1 decides to abort

  - Any of the changes might have made it to the disk

Using the log to *abort/rollback*

- UNDO(T1):
  - Go *backwards* in the *log* looking for log records belonging to T1
  - Restore the values to the old values
  - NOTE: Going backwards is important.
    - *A* was updated twice
  - In the example, we simply:
    - Restore A to 300
    - Restore B to 400
    - Restore A to 200
  - Write a log record $<T_i , X_j, V_1>$
    - such log records are called **compensation log records**
    - **<T1, A, 300>, <T1, B, 400>, <T1, A, 200>**
  - Note: No other transaction better have changed A or B in the meantime
    - *Strict two-phase locking*

Using the log to *recover*

- We don't require FORCE, so a change made by the committed transaction may not have made it to the disk before the system crashed
  - BUT, the log record did (recall our assumptions)
- REDO(T1):
  - Procedure executed to recover a committed transaction
  - E.g.
    - *<T1, START>*
    - *<T1, A, 200, 300>*
    - *<T1, B, 400, 300>*
    - *<T1, A, 300, 200>* *[[note: second update of A]]*
    - *<T1, COMMIT>*
  - By our assumptions, all the log records made it to the disk (since the transaction committed)
  - But any or none of the changes to A or B might have made it to disk

Using the log to *recover*

- REDO(T1):
  - Go *forwards* in the *log* looking for log records belonging to T1
  - Set the values to the new values
  - NOTE: Going forwards is important.
  - In the example, we simply:
    - Set A to 300
    - Set B to 300
    - Set A to 200

Idempotency

- Both redo and undo are required to *idempotent*
  - *F is idempotent, if F(x) = F(F(x)) = F(F(F(F(…F(x)))))*
- Multiple applications shouldn't change the effect
  - This is important because we don't know exactly what made it to the disk, and we can't keep track of that
  - E.g. consider a log record of the type
    - <T1, A, *incremented by 100*>
    - Old value was 200, and so new value was 300
  - But the on disk value might be 200 or 300 (since we have no control over the buffer manager)
  - So we have no idea whether to apply this log record or not
  - Hence, *value based logging* is used (also called *physical*), not operation based (also called *logical*)

Log-based recovery

- Log is maintained
- If during the normal processing, a transaction needs to abort
  - UNDO() is used for that purpose
- If the system crashes, then we need to do recovery using both UNDO() and REDO()
  - Some transactions that were going on at the time of crash may not have completed, and must be *aborted/undone*
  - Some transaction may have committed, but their changes didn't make it to disk, so they must be *redone*
  - Called *restart recovery*

Recovery Algorithm (Cont.)

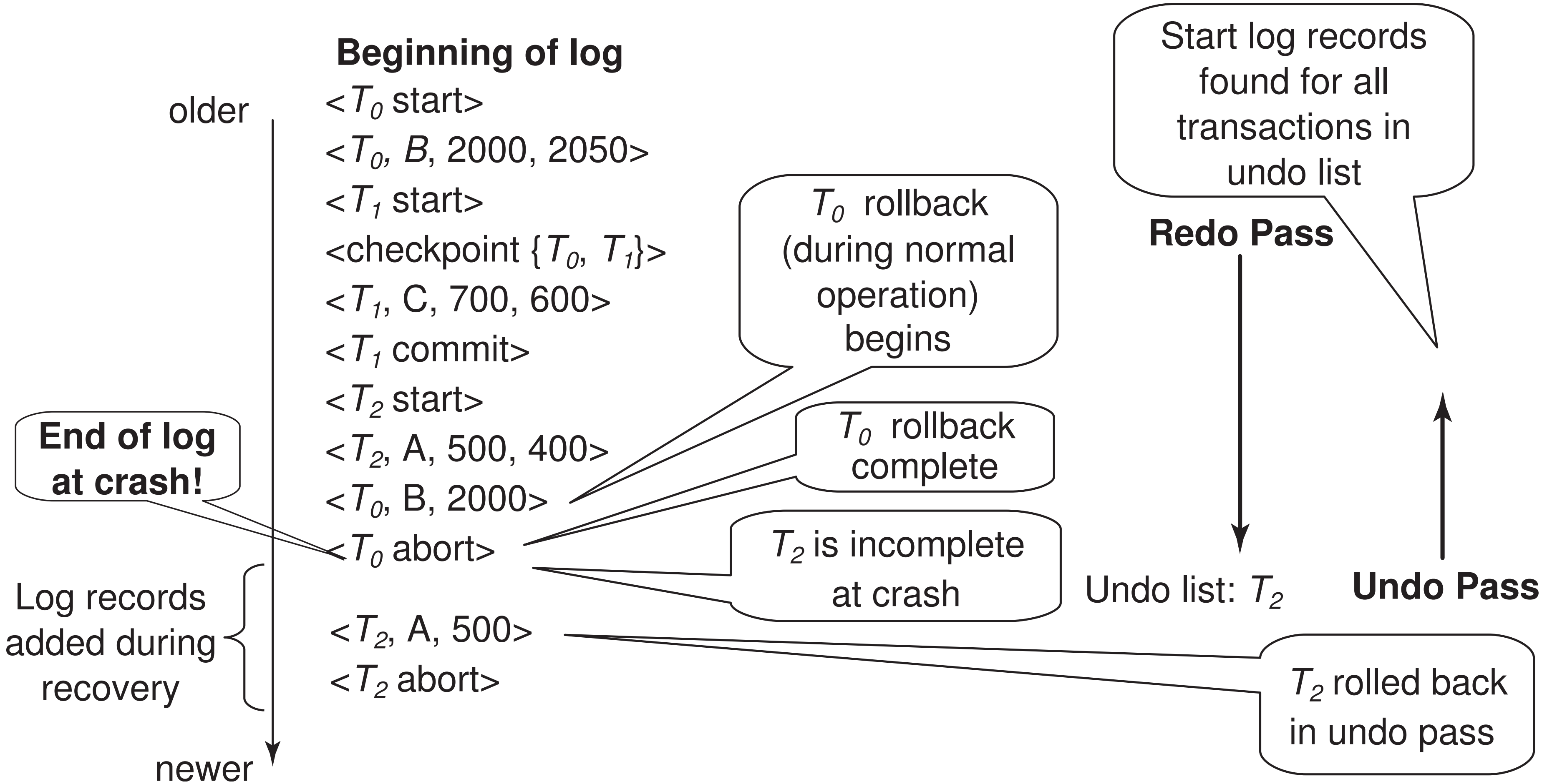
- **Recovery from failure**: Two phases
  - **Redo phase**: replay updates of **all** transactions, whether they committed, aborted, or are incomplete
  - **Undo phase**: undo all incomplete transactions
- **Redo phase**:
  1. Find last <**checkpoint** *L*> record, and set undo-list to *L*.

     - If no checkpoint record, start at the beginning
  2. Scan forward from above <**checkpoint** *L*> record
     1. Whenever a record $<T_i, X_j, V_1, V_2>$ is found, redo it by writing V_2 to X_j
     2. Whenever a log record $<T_i$ **start**$>$ is found, add T_i to undo-list
     3. Whenever a log record $<T_i$ **commit**$>$ *or* $<T_i$ **abort**$>$ is found, remove T_i from undo-list

Recovery Algorithm (Cont.)

- **Undo phase:**
  1. Scan log backwards from end
     1. Whenever a log record $<T_i, X_j, V_1, V_2>$ is found where T_i is in undo-list perform same actions as for transaction rollback:
        1. perform undo by writing V_1 to X_j.
        2. write a log record $<T_i , X_j, V_1>$
     2. Whenever a log record $<T_i$ **start**$>$ is found where T_i is in undo-list,
        1. Write a log record $<T_i$ **abort**$>$
        2. Remove T_i from undo-list
     3. Stop when undo-list is empty
        - i.e. $<T_i$ **start**$>$ has been found for every transaction in undo-list
- After undo phase completes, normal transaction processing can commence

Example of Recovery



Checkpointing

- How far should we go back in the log while constructing redo and undo lists ??
  - It is possible that a transaction made an update at the very beginning of the system, and that update never made it to disk
    - very very unlikely, but possible (because we don't do force)
  - For correctness, we have to go back all the way to the beginning of the log
  - Bad idea !!
- Checkpointing is a mechanism to reduce this

Checkpointing

- Periodically, the database system writes out everything in the memory to disk
  - Goal is to get the database in a state that we know (not necessarily consistent state)
- Steps:
  - Stop all other activity in the database system
  - Write out the entire contents of the memory to the disk
    - Only need to write updated pages, so not so bad
    - Entire === all updates, whether committed or not
  - Write out all the log records to the disk
  - Write out a special log record to disk
    - *<CHECKPOINT LIST_OF_ACTIVE_TRANSACTIONS>*
    - The second component is the list of all active transactions in the system right now
  - Continue with the transactions again

Restart Recovery w/ checkpoints

- Key difference: Only need to go back till the last checkpoint
- Steps:
  - undo_list:
    - Go back till the checkpoint as before.
    - Add all the transactions that were active at that time, and that didn't commit
      - e.g. possible that a transactions started before the checkpoint, but didn't finish till the crash
  - redo_list:
    - Similarly, go back till the checkpoint constructing the redo_list
    - Add all the transactions that were active at that time, and that did commit
  - Do UNDOs and REDOs as before

Recap so far …

- Log-based recovery
  - Uses a *log* to aid during recovery
- UNDO()
  - Used for normal transaction abort/rollback, as well as during restart recovery
- REDO()
  - Used during restart recovery
- Checkpoints
  - Used to reduce the restart recovery time

Write-ahead logging

- We assumed that log records are written to disk as soon as generated
  - Too restrictive
- Write-ahead logging:
  - Before an update on a data item (say A) makes it to disk, the log records referring to the update must be forced to disk
  - How ?
    - Each log record has a log sequence number (LSN)
      - Monotonically increasing
    - For each page in the memory, we maintain the LSN of the *last log record* that updated a record on this page
      - *pageLSN*
    - If a page *P* is to be written to disk, all the log records till *pageLSN(P)* are forced to disk

Write-ahead logging

- Write-ahead logging (WAL) is sufficient for all our purposes
  - All the algorithms discussed before work
- Note the special case:
  - A transaction is not considered committed, unless the <T, commit> record is on disk

Other issues

- The system halts during checkpointing
  - Not acceptable
  - Advanced recovery techniques allow the system to continue processing while checkpointing is going on
- System may crash during recovery
  - Our simple protocol is actually fine
  - In general, this can be painful to handle
- B+-Tree and other indexing techniques
  - Strict 2PL is typically not followed (we didn't cover this)
  - So physical logging is not sufficient; must have logical logging

Other issues

- ARIES: Considered *the canonical description of log-based recovery*
  - Used in most systems
  - Has many other types of log records that simplify recovery significantly
- Loss of disk:
  - Can use a scheme similar to checkpoining to periodically dump the database onto *tapes* or *optical storage*
  - Techniques exist for doing this while the transactions are executing (called *fuzzy dumps)*
- Shadow paging:
  - Read up

Recap

- STEAL vs NO STEAL, FORCE vs NO FORCE
  - We studied how to do STEAL and NO FORCE through log-based recovery scheme

	No Steal	Steal
No Force		Desired
Force	Trivial	

	No Steal	Steal
No Force	REDO NO UNDO	REDO UNDO
Force	NO REDO NO UNDO	NO REDO UNDO

Recap

- ACID Properties
  - Atomicity and Durability :
    - Logs, undo(), redo(), WAL etc
  - Consistency and Isolation:
    - Concurrency schemes
  - Strong interactions:
    - We had to assume Strict 2PL for proving correctness of recovery