CSE544
Data Management

Lectures 18

Transactions: Concurrency Control

Reminders

- Last lecture!
- Please fill out the course evaluation form
- Project report due by Tuesday, June 8
  No late days!

Implementing Transactions

Notice: we will discuss about ½ of these slides in class.
If you want to learn more details, the skipped slides are easy to read

Review

- What is a transaction?
- What is a schedule?
- Types:
  - Serializable
  - View serializable
  - Conflict serializable
- Types:
  - Recoverable
  - Avoid cascading aborts
  - Strict (see book)

Review

- What is a transaction?
- What is a schedule?
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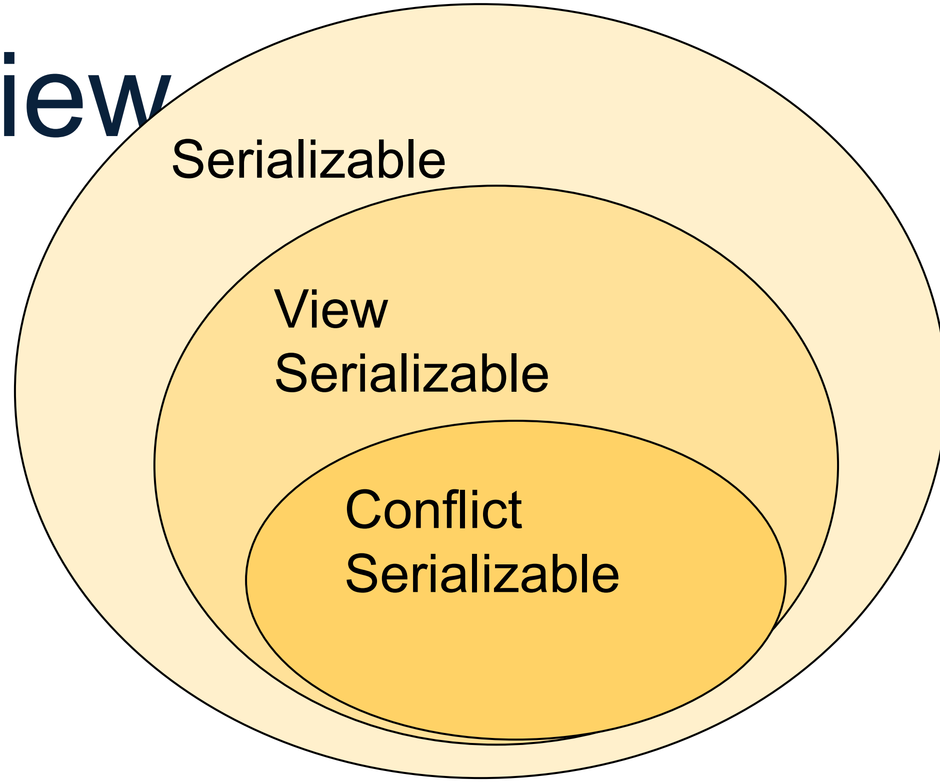
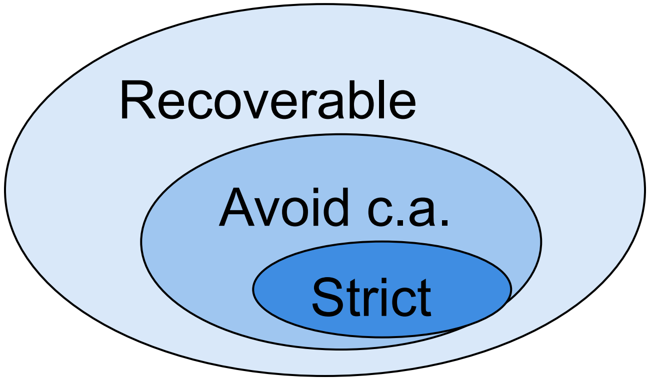
  - Serializable
  - View serializable
  - Conflict serializable
- Types:
  - Recoverable
  - Avoid cascading aborts
  - Strict (see book)





Scheduler

A.k.a. Concurrency Control Manager

- The module that schedules the transaction
- TXN T requests: READ(X) or WRITE(X),
- Scheduler answers one of:
  - Proceed
  - Put in a wait queue, schedule another TXN T'
  - Abort (!!)

Implementing a Scheduler

Two major approaches:

- Locking Scheduler
  - Aka "pessimistic concurrency control"
  - SQLite, SQL Server, DB2
- Multiversion Concurrency Control (MVCC)
  - Aka "optimistic concurrency control"
  - Postgres, Oracle: Snapshot Isolation (SI)

Lock-based Implementation of Transactions

Locking Scheduler

Simple idea:

- Each element has a unique lock
- Each transaction must first acquire the lock before reading/writing that element
- If the lock is taken, then wait
- The transaction must release the lock(s)

Actions on Locks

$L_i(A)$ = transaction T_i acquires lock for element A

$U_i(A)$ = transaction T_i releases lock for element A

Let's see this in action…

A Non-Serializable Schedule

T1	T2
READ(A)	
A := A+100	
WRITE(A)	
	READ(A)
	A := A*2
	WRITE(A)
	READ(B)
	B := B*2
	WRITE(B)
READ(B)	
B := B+100	
WRITE(B)	

Example

T1	T2
$L_1(A)$; READ(A)	
A := A+100	
WRITE(A); $U_1(A)$; $L_1(B)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A); $U_2(A)$;
	$L_2(B)$; BLOCKED…
READ(B)	
B := B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(B)$;

Example

T1	T2
$L_1(A)$; READ(A)	
A := A+100	
WRITE(A); $U_1(A)$; $L_1(B)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A); $U_2(A)$;
	$L_2(B)$; BLOCKED…
READ(B)	
B := B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(B)$;

Schedule is conflict-serializable

But…

T1	T2
$L_1(A)$; READ(A)	
A := A+100	
WRITE(A); $U_1(A)$;	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A); $U_2(A)$;
	$L_2(B)$; READ(B)
	B := B*2
	WRITE(B); $U_2(B)$;
$L_1(B)$; READ(B)	
B := B+100	
WRITE(B); $U_1(B)$;	

But…

T1	T2
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	A := A*2
	WRITE(A); $U_2(A)$;
	$L_2(B)$; READ(B)
	B := B*2
	WRITE(B); $U_2(B)$;
$L_1(B)$; READ(B)	
B := B+100	
WRITE(B); $U_1(B)$;	

Locks did not enforce conflict-serializability !!! What's wrong ?

Two Phase Locking (2PL)

The 2PL rule:

In every transaction, all lock requests must precede all unlock requests

Example: 2PL transactions

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A := A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B := B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;

Conflict-serializable

Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

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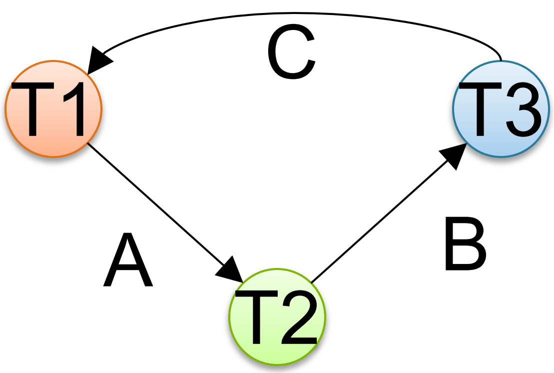
Proof. Suppose not: then there exists a cycle in the precedence graph.



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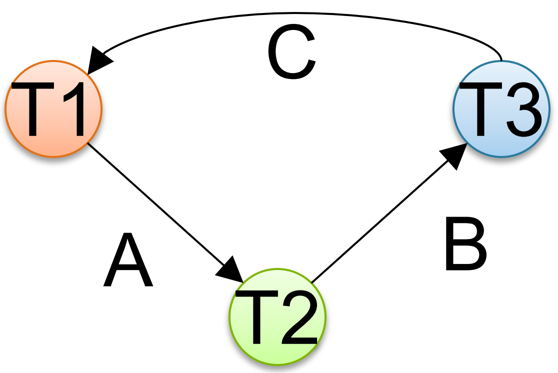

Then there is the following **temporal** cycle in the schedule:

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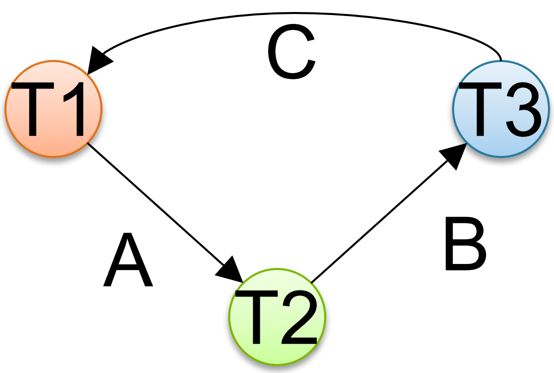

Then there is the following **temporal** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$ why?

$U_1(A)$ happened strictly *before* $L_2(A)$

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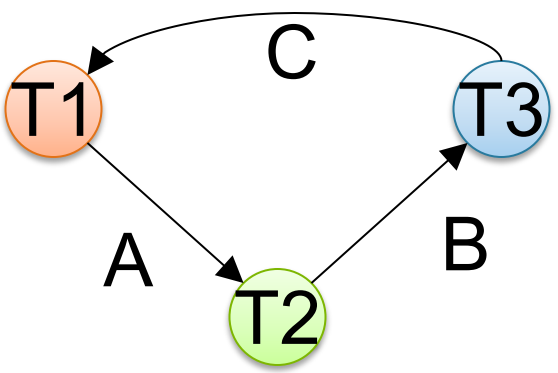

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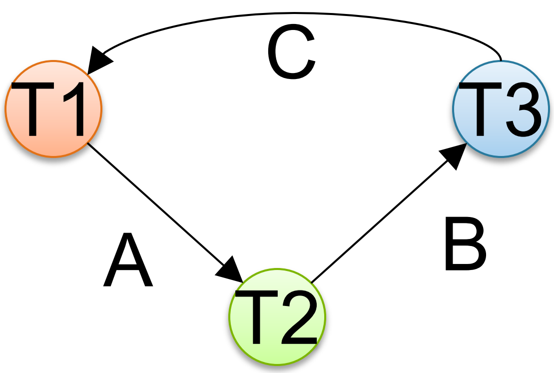

Then there is the following **temporal** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$

$L_2(A) \rightarrow U_2(B)$ why?

$L_2(A)$ happened strictly *before* $U_1(A)$

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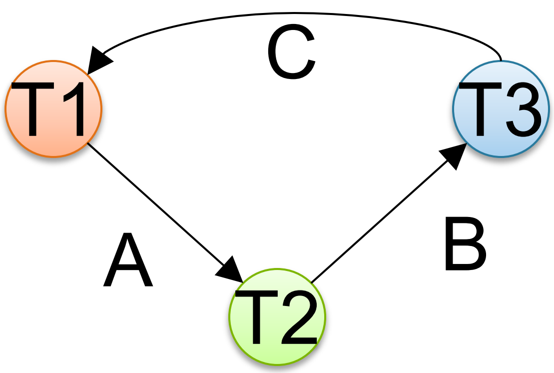

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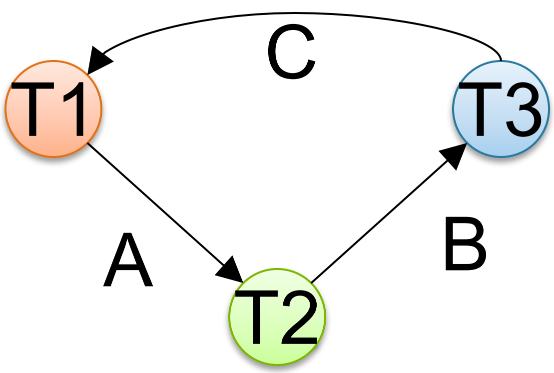

Then there is the following **temporal** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$

$L_2(A) \rightarrow U_2(B)$

$U_2(B) \rightarrow L_3(B)$

why?

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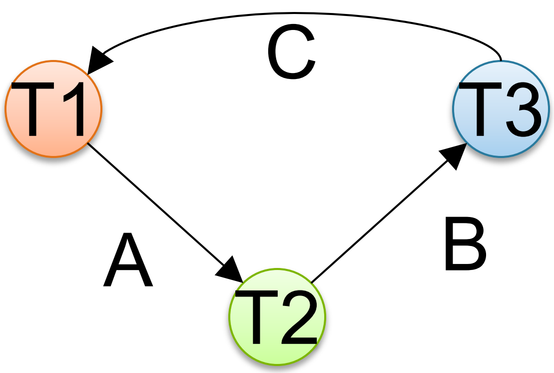

Then there is the following **temporal** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$

$L_2(A) \rightarrow U_2(B)$

$U_2(B) \rightarrow L_3(B)$

......etc.....

Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

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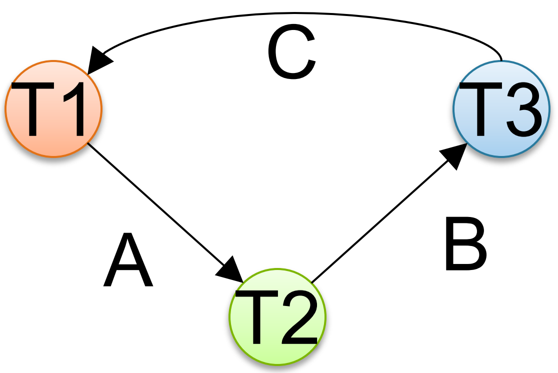


Then there is the following **temporal** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$
$L_2(A) \rightarrow U_2(B)$
$U_2(B) \rightarrow L_3(B)$
$L_3(B) \rightarrow U_3(C)$
$U_3(C) \rightarrow L_1(C)$
$L_1(C) \rightarrow U_1(A)$

Cycle in time: Contradiction

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

Non-recoverable schedule

Strict 2PL

The Strict 2PL rule:

All locks are held until commit/abort:
All unlocks are done *with* commit/abort.

Strict 2PL

T1	T2
$L_1(A)$; READ(A)	
A :=A+100	
WRITE(A);	
	$L_2(A)$; BLOCKED…
$L_1(B)$; READ(B)	
B :=B+100	
WRITE(B);	
Rollback & $U_1(A)$;$U_1(B)$;	
	…GRANTED; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; READ(B)
	B := B*2
	WRITE(B);
	Commit & $U_2(A)$; $U_2(B)$;

Strict 2PL

- Lock-based systems always use strict 2PL
- Easy to implement:
  - When TXN requests READ(X) or WRITE(X), insert a lock requests on X
  - When the transaction commits/aborts, release all locks
- Conflict-serializable
- Strict
  - Thus: avoids-cascading aborts

Another problem: Deadlocks

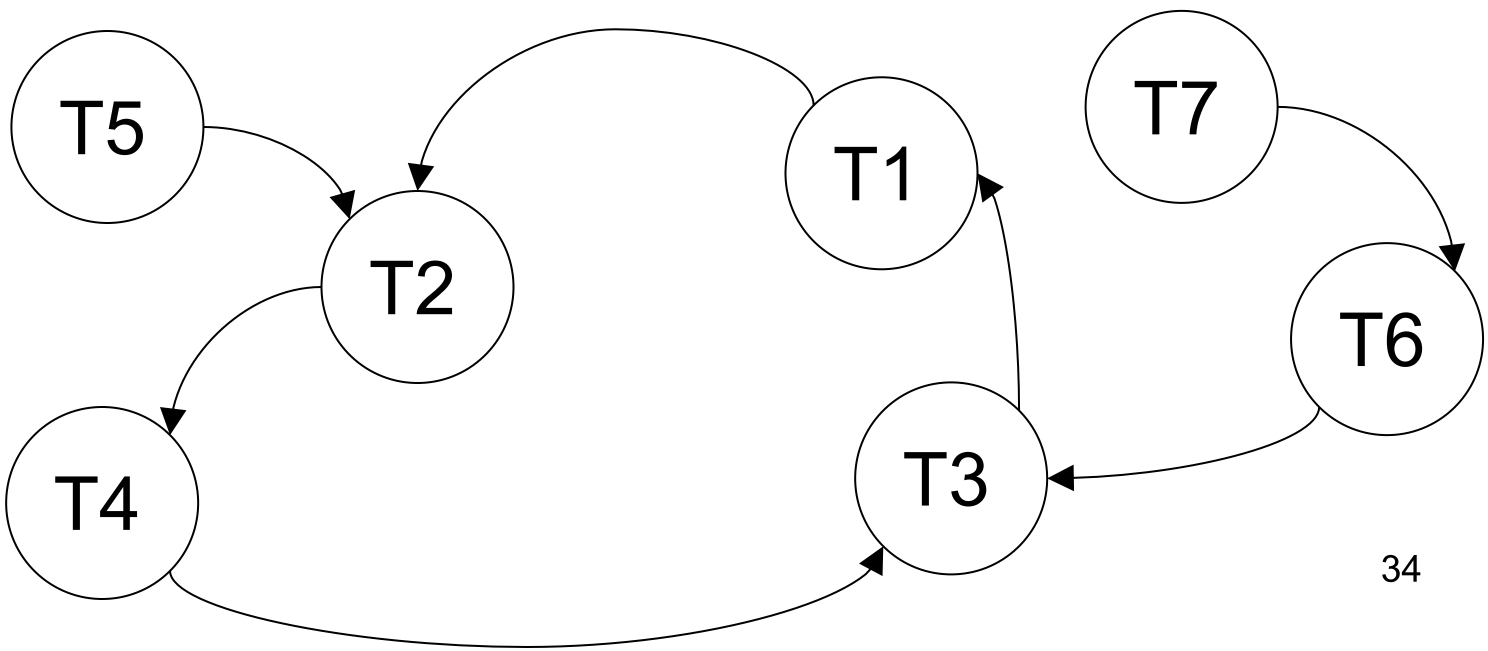
- T_1: R(A), W(B)
- T_2: R(B), W(A)

- T_1 holds the lock on A, waits for B
- T_2 holds the lock on B, waits for A

This is a deadlock!

Another problem: Deadlocks

- Deadlock = when *waits-for* graph has a cycle
- Check the graph periodically; if deadlock is detected then pick a txn T and abort it; recheck more often.



Lock Modes

- S = shared lock (for READ)
- X = exclusive lock (for WRITE)

Lock compatibility matrix:

	None	S	X
None			
S			
X			

Lock Modes

- S = shared lock (for READ)
- X = exclusive lock (for WRITE)

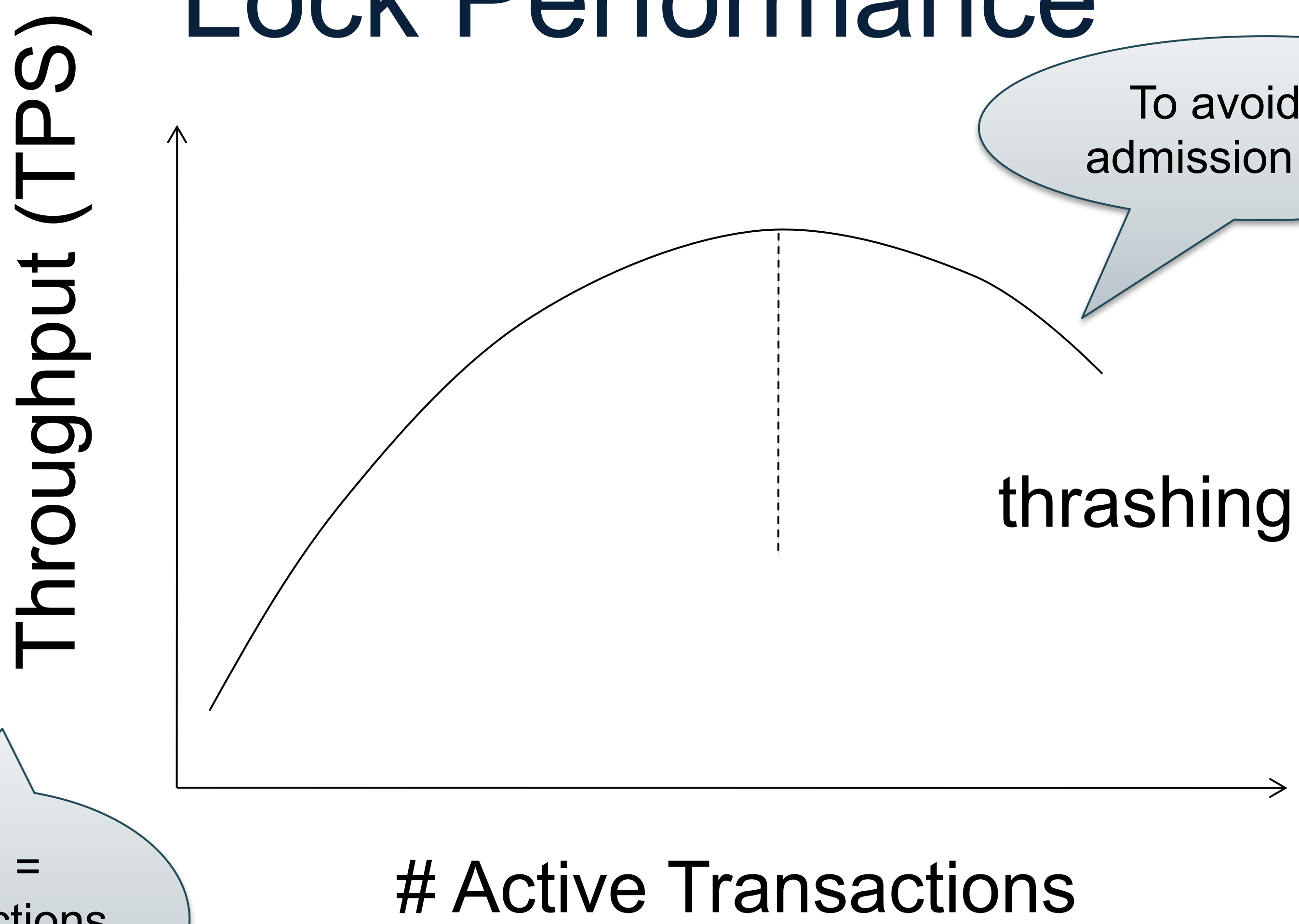
Lock compatibility matrix:

	None	S	X
None	✓	✓	✓
S	✓	✓	✗
X	✓	✗	✗

Lock Granularity

- Fine granularity locking (e.g., tuples)
  - High concurrency
  - High overhead in managing locks
  - E.g., SQL Server
- Coarse grain locking (e.g., tables, entire database)
  - Many false conflicts
  - Less overhead in managing locks
  - E.g., SQL Lite
- Solution: lock escalation changes granularity as needed

Lock Performance



To avoid, use admission control

TPS = Transactions per second

Optimistic concurrency control

Optimistic CC

- Proceeds more aggressively, but in case of conflicts are more likely to require abort
- Three main abstractions:
  – Timestamps
  – Multiversions
  – Validation
- Will illustrate them separately

Timestamps

Timestamps

- Each transaction receives a unique timestamp TS(T)

Could be:

- The system's clock
- A unique counter, incremented by the scheduler

Timestamps

Main invariant:

The timestamp order defines
the serialization order of the transaction

Will generate a schedule that is view-equivalent to a serial schedule, and strict

Timestamps

With each element X, associate

- RT(X) = the highest timestamp of any transaction U that read X
- WT(X) = the highest timestamp of any transaction U that wrote X
- C(X) = the commit bit: true when transaction with highest timestamp that wrote X committed

Warning

Confusing notation:

- $r_T(X)$ = txn T reads element X
- RT(X) = the “read timestamp” of X
- TS(T) = the ”timestamp” of txn T

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$$w_U(X) \ . \ . \ . \ r_T(X)$$

Should we allow this?

Main Idea

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Should we allow this?

Suppose the history was:

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
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- Consider these cases:

$$w_U(X) \;.\;.\;.\; r_T(X)$$

Should we allow this?

Suppose the history was:

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$w_U(X) \ . \ . \ . \ r_T(X)$

Should we allow this?

Suppose the history was:

OK

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$

$WT(X) \leq TS(T)$

Main Idea

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- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$$w_U(X) \ldots r_T(X)$$

Should we allow this?

Suppose the history was:

OK

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$

START(T), ...,START(U), ..., $w_U(X)$, ..., $r_T(X)$

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$$w_U(X) \; . \; . \; . \; r_T(X)$$

Should we allow this?

Suppose the history was:

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$

OK

START(T), ...,START(U), ..., $w_U(X)$, ..., $r_T(X)$

Too late

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$$w_U(X) \;.\;.\;.\; r_T(X)$$

Should we allow this?

Suppose the history was:

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$ — OK

START(T), ...,START(U), ..., $w_U(X)$, ..., $r_T(X)$ — Too late

$WT(X) > TS(T)$

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$$w_U(X) \ldots r_T(X)$$
$$r_U(X) \ldots w_T(X)$$
$$w_U(X) \ldots w_T(X)$$

Should we allow this?

- Similarly for the other cases

Details

Read too late:

- T wants to read X, and WT(X) > TS(T)

START(T) … START(U) … $w_U(X)$ . . . $r_T(X)$

Need to rollback T !

Details

Write too late:

- T wants to write X, and RT(X) > TS(T)

START(T) … START(U) … $r_U(X)$ . . . $w_T(X)$

Need to rollback T !

Details

Write too late, but we can still handle it:

- T wants to write X, and
  $RT(X) \leq TS(T)$ but $WT(X) > TS(T)$

START(T) … START(V) … $w_V(X)$ . . . $w_T(X)$

Don’t write X at all !
(Thomas’ rule)

Simplified TS

$w_U(X) \ldots r_T(X)$
$r_U(X) \ldots w_T(X)$
$w_U(X) \ldots w_T(X)$

Only for transactions that do not abort

Otherwise, may result in non-recoverable schedule

Request is $r_T(X)$
?

Request is $w_T(X)$
?

Simplified TS

$w_U(X) \ldots r_T(X)$
$r_U(X) \ldots w_T(X)$
$w_U(X) \ldots w_T(X)$

Only for transactions that do not abort

Otherwise, may result in non-recoverable schedule

Request is $r_T(X)$
If WT(X) > TS(T) then ROLLBACK
Else READ and update RT(X) to larger of TS(T) or RT(X)

Request is $w_T(X)$
?

Simplified TS

$w_U(X) \ldots r_T(X)$
$r_U(X) \ldots w_T(X)$
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Only for transactions that do not abort

Otherwise, may result in non-recoverable schedule

Request is $r_T(X)$
 If WT(X) > TS(T) then ROLLBACK
 Else READ and update RT(X) to larger of TS(T) or RT(X)

Request is $w_T(X)$
 If RT(X) > TS(T) then ROLLBACK
 Else if WT(X) > TS(T) ignore write & continue (Thomas Write Rule)
 Otherwise, WRITE and update WT(X) =TS(T)

Simplified TS

- **Fact**: the simplified timestamp-based scheduling with Thomas' rule ensures that the schedule is view-serializable

Full TS

- Use the commit bit C(X) to keep track if the transaction that last wrote X has committed

Full TS

Read dirty data:

- T wants to read X, and WT(X) < TS(T)
- Seems OK, but…

START(U) … START(T) … $w_U(X)$. . . $r_T(X)$... ABORT(U)

If C(X)=false, T needs to wait for it to become true

Full TS

Thomas' rule needs to be revised:

- T wants to write X, and WT(X) > TS(T)
- Seems OK not to write at all, but …

START(T) … START(U)… w_U(X). . . w_T(X)… ABORT(U)

If C(X)=false, T needs to wait for it to become true

Full TS

Request is $r_T(X)$
If WT(X) > TS(T) then ROLLBACK
Else If C(X) = false, then WAIT
Else READ and update RT(X) to larger of TS(T) or RT(X)

Request is $w_T(X)$
If RT(X) > TS(T) then ROLLBACK
Else if WT(X) > TS(T)
Then If C(X) = false then WAIT
else IGNORE write (Thomas Write Rule)
Otherwise, WRITE, and update WT(X)=TS(T), C(X)=false

Full TS

- Fact: full timestamp-based scheduling is view-serializable and avoids cascasing aborts

Timestamps

Main takeaway:

- TS defines the serialization order
- Simplifies the scheduler:
  - If action is consistent with serialization order, then proceed
  - Otherwise, ABORT

Multiversions

Multiversion Timestamp

- When transaction T requests r(X) but WT(X) > TS(T), then T must rollback
- Idea: keep multiple versions of X: X_t, X_{t-1}, X_{t-2}, . . .

$$TS(X_t) > TS(X_{t-1}) > TS(X_{t-2}) > \ldots$$

- Let T read an older version, with appropriate timestamp

Details

- When $w_T(X)$ occurs,
  create a new version, denoted X_t where t = TS(T)

- When $r_T(X)$ occurs,
  find most recent version X_t such that t <= TS(T)
  Notes:
  - $WT(X_t)$ = t and it never changes
  - $RT(X_t)$ must still be maintained to check legality of writes

- Can delete X_t if we have a later version X_{t1} and all active transactions T have TS(T) > t1

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{18}

$R_6(X)$ -- what happens?

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens?

$W_5(X)$ – what happens?

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{18}

$R_6(X)$ -- what happens? Return X_3

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens?

$W_5(X)$ – what happens?

When can we delete X_3?

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$R_6(X)$ -- what happens? Return X_3

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens? Return X_{14}

$W_5(X)$ – what happens?

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

$R_6(X)$ -- what happens? Return X_3

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens? Return X_{14}

$W_5(X)$ – what happens?

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

$R_6(X)$ -- what happens? Return X_3
$W_{14}(X)$ – what happens?
$R_{15}(X)$ – what happens? Return X_{14}
$W_5(X)$ – what happens? ABORT

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

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Example (in class)

TS(T)=6

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$R_6(X)$ -- what happens? Return X_3
$W_{14}(X)$ – what happens?
$R_{15}(X)$ – what happens? Return X_{14}
$W_5(X)$ – what happens? ABORT

When can we delete X_3? When min $TS(T) \geq 9$

Multiversion

Takeaways:

- Reduces the number of aborts due to late reads
- Simplifies rollback
- Handles “phantoms”

Validation

Concurrency Control by Validation

- TXN reads elements, performs all updates on local copies

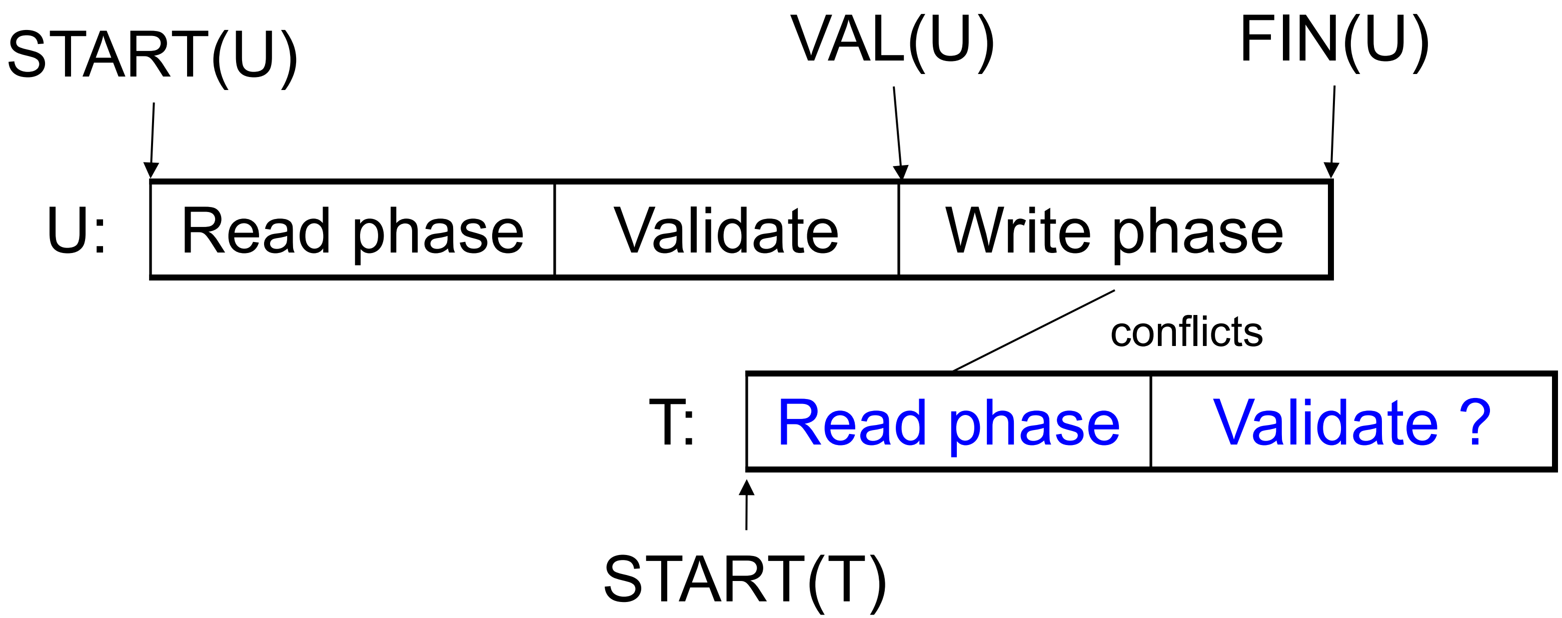
- At commit time:
  - CC manager performs *<u>validation</u>*
  - If OK, then it writes the local copies to disk
  - If not OK then aborts

Concurrency Control by Validation

- Each transaction T defines:
  - a *read set* RS(T) and
  - a *write set* WS(T)
- Each TXN has three phases:
  - Read elements RS(T): Time = START(T)
  - Validate: Time = VAL(T)
  - Writes elements WS(T). Time = FIN(T)

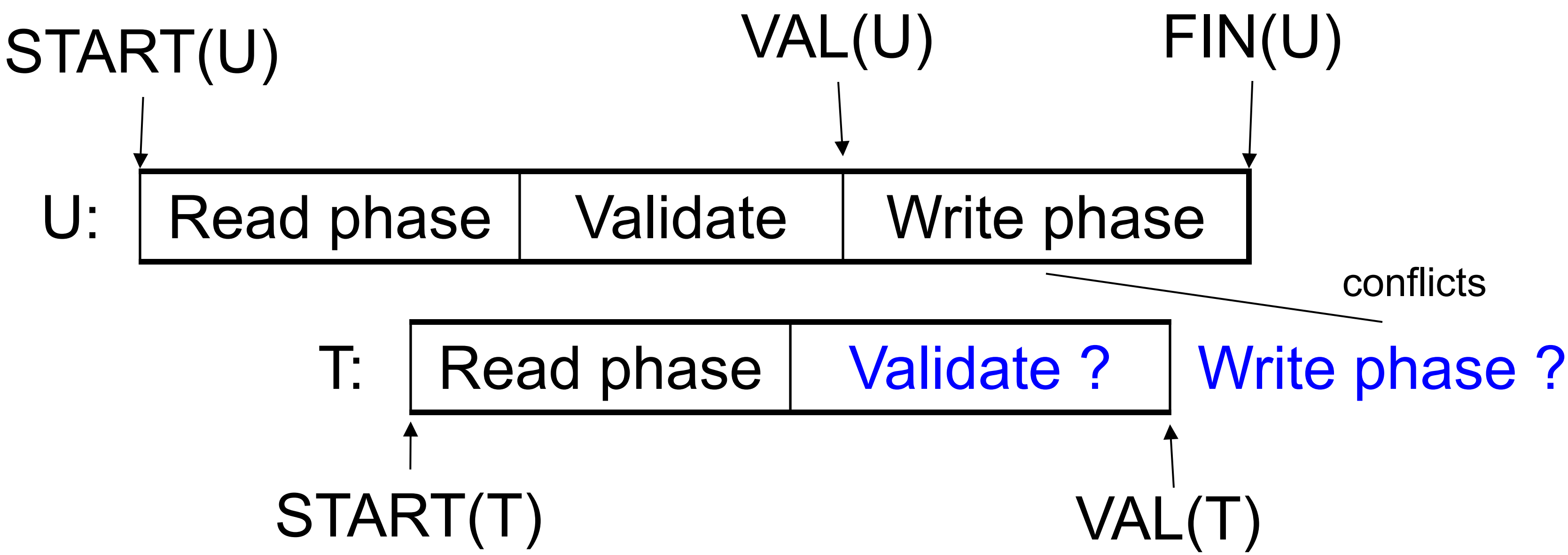
Main invariant: the serialization order is VAL(T)

Avoid $r_T(X)$ - $w_U(X)$ Conflicts



IF RS(T) ∩ WS(U) and FIN(U) > START(T)
Then ROLLBACK(T)

Avoid $w_T(X)$ - $w_U(X)$ Conflicts



IF WS(T) $\cap$ WS(U) and FIN(U) > VAL(T)
Then ROLLBACK(T)

Validation

Takeaways:

- READs/WRITEs proceed without delay
- Only delay happens at validation time
- May abort aggressively

Snapshot Isolation (SI)

A variant of multiversion/validation

- Very efficient, and very popular
- Oracle, PostgreSQL, SQL Server 2005

Warning: not serializable

- Earlier versions of postgres implemented SI for the SERIALIZABLE isolation level
- Extension of SI to serializable has been implemented recently
- Will discuss only the standard SI (non-serializable)

Snapshot Isolation Rules

- Each transactions receives a timestamp TS(T)
- Transaction T sees snapshot at time TS(T) of the database
- When T commits, updated pages are written to disk
- Write/write conflicts resolved by “first committer wins” rule
  - Loser gets aborted
- Read/write conflicts are ignored

Snapshot Isolation (Details)

- Multiversion concurrency control:
  - Versions of X: X_{t1}, X_{t2}, X_{t3}, . . .
- When T reads X, return X_t,
  where t is max s.t. $t \le TS(T)$
- When T writes X:
  if other transaction updated X, abort

What Works and What Not

- No dirty reads (Why ?)
- No inconsistent reads (Why ?)
  - A: Each transaction reads a consistent snapshot

- No lost updates ("first committer wins")

- Moreover: no reads are ever delayed

- However: read-write conflicts not caught ! "Write skew"

Write Skew

Invariant: $X + Y \geq 0$

```
T1:
  READ(X);
  if X >= 50
      then Y = -50; WRITE(Y)
  COMMIT
```

```
T2:
  READ(Y);
  if Y >= 50
      then X = -50; WRITE(X)
  COMMIT
```

In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Write Skew

Invariant: $X + Y \geq 0$

```
T1:
  READ(X);
  if X >= 50
       then Y = -50; WRITE(Y)
  COMMIT
```

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T2:
  READ(Y);
  if Y >= 50
       then X = -50; WRITE(X)
  COMMIT
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In our notation: $R_1(X)$, $R_2(Y)$, $W_1(Y)$, $W_2(X)$, C_1,C_2

X_0 Y_0

Write Skew

Invariant: $X + Y \geq 0$

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T1:
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In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

$X_0 \quad Y_0$

Write Skew

Invariant: $X + Y \geq 0$

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T1:
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  COMMIT
```

In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Should have aborted T1, but SI doesn't keep RT(Y)

$X_0 \quad Y_0 \quad Y_1$

Write Skew

Invariant: $X + Y \geq 0$

```
T1:
  READ(X);
  if X >= 50
      then Y = -50; WRITE(Y)
  COMMIT
```

```
T2:
  READ(Y);
  if Y >= 50
      then X = -50; WRITE(X)
  COMMIT
```

In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Should have aborted T1, but SI doesn't keep RT(Y)

$X_0 \quad Y_0 \quad Y_1 \quad X_2$

Write Skew

Invariant: $X + Y \geq 0$

T1:	T2:
READ(X); if X >= 50 then Y = -50; WRITE(Y) COMMIT	READ(Y); if Y >= 50 then X = -50; WRITE(X) COMMIT

In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Should have aborted T1, but SI doesn't keep RT(Y)

$X_0 \quad Y_0 \quad Y_1 \quad X_2$

Write Skew

Invariant: $X + Y \geq 0$

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T1:
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  if Y >= 50
      then X = -50; WRITE(X)
  COMMIT
```

In our notation: $R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Should have aborted T1, but SI doesn't keep RT(Y)

$X_0 \quad Y_0 \quad Y_1 \quad X_2$

Starting with X=50,Y=50, we end with X=-50, Y=-50.
Non-serializable !!!

Discussions

- Snapshot isolation (SI) is like repeatable reads but also avoids some (not all) phantoms
- If DBMS runs SI and the app needs serializable:
  - use dummy writes for all reads to create write-write conflicts… but that is confusing for developers
- Extension of SI to make it serializable is implemented in postgres

Phantom Problem

- So far we have assumed the database to be a *static* collection of elements (=tuples)

- If tuples are inserted/deleted then the *phantom problem* appears

Suppose there are two blue products, A1, A2:

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('A3','blue')
SELECT * FROM Product WHERE color='blue'	

Is this schedule serializable ?

Suppose there are two blue products, A1, A2:

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('A3','blue')
SELECT * FROM Product WHERE color='blue'	

Is this schedule serializable ?

No: T1 sees a "phantom" product A3

Suppose there are two blue products, A1, A2:

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('A3','blue')
SELECT * FROM Product WHERE color='blue'	

$R_1(A1);R_1(A2);W_2(A3);R_1(A1);R_1(A2);R_1(A3)$

Suppose there are two blue products, A1, A2:

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color=‘blue’	
	INSERT INTO Product(name, color) VALUES (‘A3’,’blue’)
SELECT * FROM Product WHERE color=‘blue’	

$R_1(A1);R_1(A2);W_2(A3);R_1(A1);R_1(A2);R_1(A3)$

$W_2(A3);R_1(A1);R_1(A2);R_1(A1);R_1(A2);R_1(A3)$

Suppose there are two blue products, A1, A2:

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('A3','blue')
SELECT * FROM Product WHERE color='blue'	

But this is conflict-serializable!

$R_1(A1);R_1(A2);W_2(A3);R_1(A1);R_1(A2);R_1(A3)$

$W_2(A3);R_1(A1);R_1(A2);R_1(A1);R_1(A2);R_1(A3)$

Phantom Problem

- A "phantom" is a tuple that is invisible during part of a transaction execution but not invisible during the entire execution

- In our example:
  - T1: reads list of products
  - T2: inserts a new product
  - T1: re-reads: a new product appears !

Phantom Problem

- In a ***static*** database:
  - Conflict serializability implies serializability

- In a ***dynamic*** database, this may fail due to phantoms

- Strict 2PL guarantees conflict serializability, but not serializability

Dealing With Phantoms

- Lock the entire table
- Lock the index entry for ‘blue’
  - If index is available
- Or use predicate locks
  - A lock on an arbitrary predicate

Dealing with phantoms is expensive !

Summary of Serializability

- Serializable schedule = equivalent to a serial schedule
- (strict) 2PL guarantees *conflict serializability*
  - What is the difference?
- **Static database**:
  - *Conflict serializability* implies serializability
- **Dynamic database**:
  - *Conflict serializability* plus *phantom management* implies serializability

Weaker Isolation Levels

- Serializable are expensive to implement

- SQL allows the application to choose a more efficient implementation, which is not always serializable: *weak isolation levels*

Isolation Levels in SQL

1. "Dirty reads"
   SET TRANSACTION ISOLATION LEVEL READ UNCOMMITTED

2. "Committed reads"
   SET TRANSACTION ISOLATION LEVEL READ COMMITTED

3. "Repeatable reads"
   SET TRANSACTION ISOLATION LEVEL REPEATABLE READ

4. Serializable transactions
   SET TRANSACTION ISOLATION LEVEL SERIALIZABLE

ACID

Lost Update

Write-Write Conflict

T_1	T_2
T_1: READ(A)	
	T_2: READ(A);
T_1: A := A+5	
	T_2: A := A*1.3
T_1: WRITE(A)	
	T_2: WRITE(A);

Never allowed at any level

1. Isolation Level: Dirty Reads

- "Long duration" WRITE locks
  - Strict 2PL
- No READ locks
  - Read-only transactions are never delayed

Possible problems: dirty and inconsistent reads

1. Isolation Level: Dirty Reads

Write-Read Conflict

T_1: WRITE(A)

T_2: READ(A)

T_1: ABORT

1. Isolation Level: Dirty Reads

Write-Read Conflict

T_1: A := 20; B := 20;	
T_1: WRITE(A)	
	T_2: READ(A);
	T_2: READ(B);
T_1: WRITE(B)	

Inconsistent read

2. Isolation Level: Read Committed

- "Long duration" WRITE locks
  - Strict 2PL
- "Short duration" READ locks
  - Only acquire lock while reading (not 2PL)

Unrepeatable reads:
When reading same element twice, may get two different values

2. Isolation Level: Read Committed

Read-Write Conflict

T_2: READ(A);

T_1: WRITE(A)
COMMIT

T_2: READ(A);

Unrepeatable read

3. Isolation Level: Repeatable Read

- "Long duration" WRITE locks
  - Strict 2PL
- "Long duration" READ locks
  - Strict 2PL

Why ?

This is not serializable yet !!!

4. Isolation Level Serializable

- "Long duration" WRITE locks
  - Strict 2PL
- "Long duration" READ locks
  - Strict 2PL
- Predicate locking
  - To deal with phantoms

Beware!

In commercial DBMSs:

- Default level may not be serializable
- Default level differs between DBMSs
- Some engines support subset of levels!
- Also, some DBMSs do NOT use locking and different isolation levels can lead to different pbs

Bottom line: Read the doc for your DBMS!

Final Thoughts on Transactions

- Benchmarks: TPC/C; typical throughput: x100's TXN/second
- New trend: multicores
  - Current technology can scale to x10's of cores, but not beyond!
  - Major bottleneck: latches that serialize the cores
- New trend: distributed TXN
  - NoSQL: give up serialization
  - Serializable: very difficult e.g.Spanner w/ Paxos