Concurrency Control

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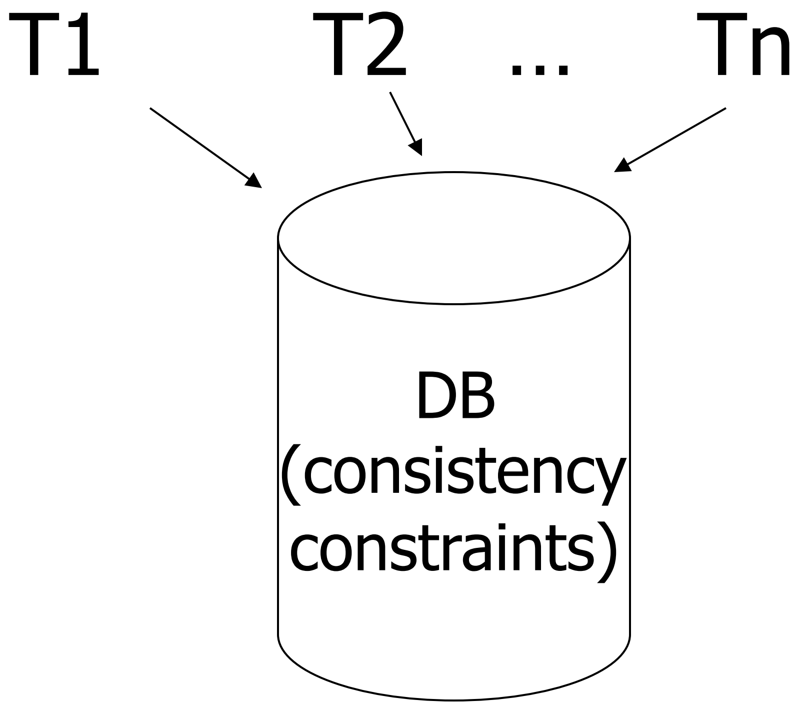
Thanks to

- These slides are authored by Hector Garcia Molina (Stanford), 2002.
- They follow the class textbook ("Stanford").

Chapter 18 [18] Concurrency Control



Example:

```
T1:  Read(A)
     $A \leftarrow A+100$
     Write(A)
     Read(B)
     $B \leftarrow B+100$
     Write(B)
```

```
T2:  Read(A)
     $A \leftarrow A \times 2$
     Write(A)
     Read(B)
     $B \leftarrow B \times 2$
     Write(B)
```

Constraint: $A=B$

Schedule A

T1	T2
Read(A); A ← A+100	
Write(A);	
Read(B); B ← B+100;	
Write(B);	
	Read(A);A ← A×2;
	Write(A);
	Read(B);B ← B×2;
	Write(B);

Schedule A

T1	T2	A	B
		25	25
Read(A); A ← A+100			
Write(A);		125	
Read(B); B ← B+100;			
Write(B);			125
	Read(A);A ← A×2;		
	Write(A);	250	
	Read(B);B ← B×2;		
	Write(B);		250
		250	250

Schedule B

T1	T2
	Read(A);A ← A×2;
	Write(A);
	Read(B);B ← B×2;
	Write(B);
Read(A); A ← A+100	
Write(A);	
Read(B); B ← B+100;	
Write(B);	

Schedule B

T1	T2	A	B
		25	25
	Read(A);A ← A×2;		
	Write(A);	50	
	Read(B);B ← B×2;		
	Write(B);		50
Read(A); A ← A+100			
Write(A);		150	
Read(B); B ← B+100;			
Write(B);			150
		150	150

Schedule C

T1	T2
Read(A); A ← A+100	
Write(A);	
	Read(A);A ← A×2;
	Write(A);
Read(B); B ← B+100;	
Write(B);	
	Read(B);B ← B×2;
	Write(B);

Schedule C

T1	T2	A	B
		25	25
Read(A); A ← A+100			
Write(A);		125	
	Read(A);A ← A×2;		
	Write(A);	250	
Read(B); B ← B+100;			
Write(B);			125
	Read(B);B ← B×2;		
	Write(B);		250
		250	250

Schedule D

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);	

Schedule D

T1	T2	A	B
		25	25
Read(A); A ← A+100			
Write(A);		125	
	Read(A);A ← A×2;		
	Write(A);	250	
	Read(B);B ← B×2;		
	Write(B);		50
Read(B); B ← B+100;			
Write(B);			150
		250	150

Schedule E

Same as Schedule D but with new T2’

T1	T2’
Read(A); A ← A+100	
Write(A);	
	Read(A);A ← A×1;
	Write(A);
	Read(B);B ← B×1;
	Write(B);
Read(B); B ← B+100;	
Write(B);	

Schedule E

Same as Schedule D
but with new T2′

T1	T2′	A	B
		25	25
Read(A); A ← A+100			
Write(A);		125	
	Read(A);A ← A×1;		
	Write(A);	125	
	Read(B);B ← B×1;		
	Write(B);		25
Read(B); B ← B+100;			
Write(B);			125
		125	125

- Want schedules that are “good”, regardless of
  - initial state and
  - transaction semantics
- Only look at order of read and writes

Example:

$Sc=r_1(A)w_1(A)r_2(A)w_2(A)r_1(B)w_1(B)r_2(B)w_2(B)$

Example:

$$Sc=r_1(A)w_1(A)\underbrace{r_2(A)w_2(A)}\underbrace{r_1(B)w_1(B)}r_2(B)w_2(B)$$

$$Sc'=\underbrace{r_1(A)w_1(A)\ r_1(B)w_1(B)}_{T_1}\underbrace{r_2(A)w_2(A)r_2(B)w_2(B)}_{T_2}$$

The Transaction Game

A								
B								
T1								
T2								

The Transaction Game

A	r	w	r	w				
B					r	w	r	w
T1	r	w			r	w		
T2			r	w			r	w

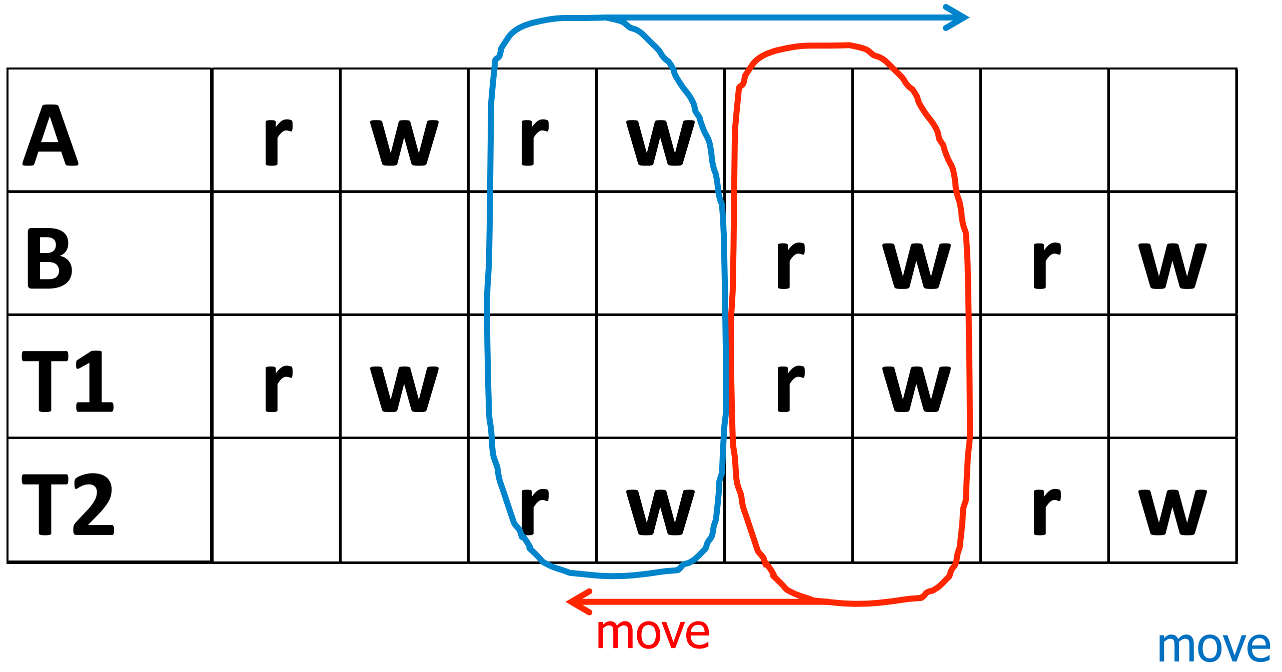
The Transaction Game

A	r	w	r	w				
B					r	w	r	w
T1	r	w			r	w		
T2			r	w			r	w

until column

hits something

can move column



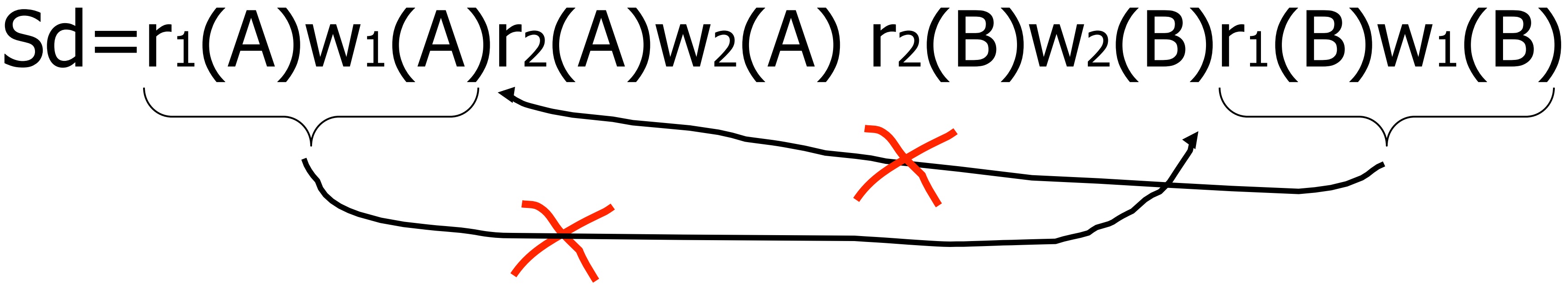
A	r	w	r	w				
B					r	w	r	w
T1	r	w			r	w		
T2			r	w			r	w

A	r	w			r	w		
B			r	w			r	w
T1	r	w	r	w				
T2					r	w	r	w

Schedule D

A	**r**	**w**	**r**	**w**				
B					**r**	**w**	**r**	**w**
T1	**r**	**w**					**r**	**w**
T2			**r**	**w**	**r**	**w**		

However, for Sd:



$Sd = \underbrace{r_1(A)w_1(A)}r_2(A)w_2(A)\ r_2(B)w_2(B)\underbrace{r_1(B)w_1(B)}$

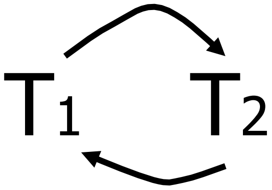
- as a matter of fact,

  T_2 must precede T_1

  in any equivalent schedule,

  i.e., $T_2 \rightarrow T_1$

- $T_2 \rightarrow T_1$
- Also, $T_1 \rightarrow T_2$

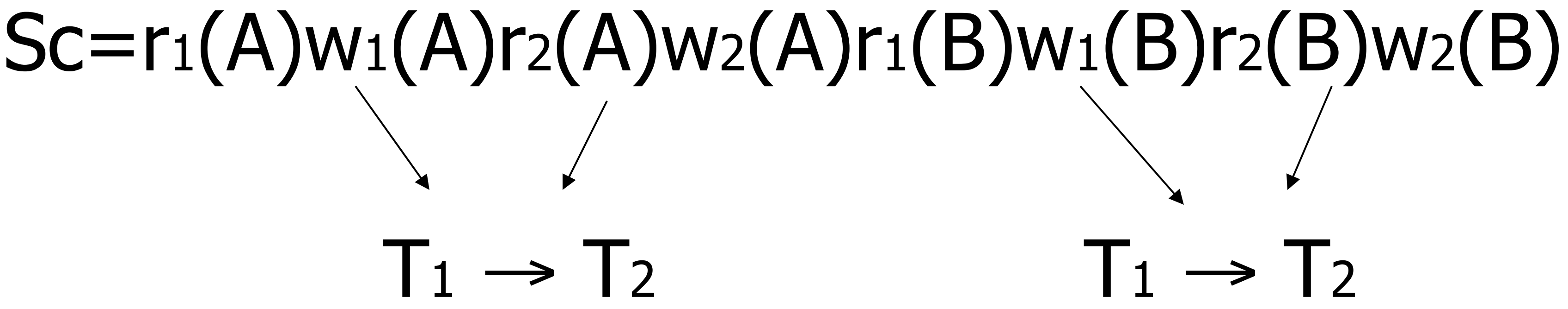


⇨ Sd cannot be rearranged
into a serial schedule

⇨ Sd is not “equivalent” to
any serial schedule

⇨ Sd is “bad”

Returning to Sc

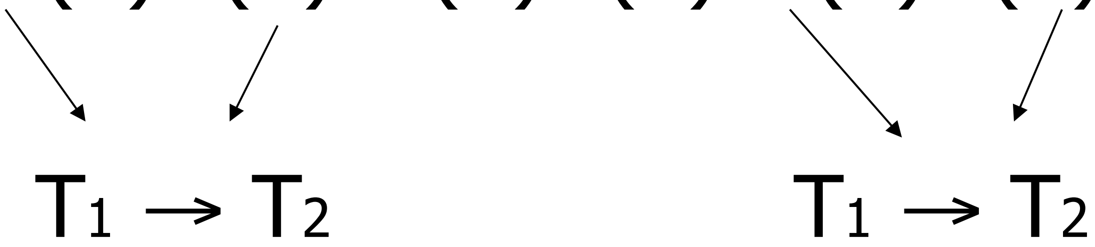


$Sc=r_1(A)w_1(A)r_2(A)w_2(A)r_1(B)w_1(B)r_2(B)w_2(B)$

$T_1 \rightarrow T_2$ $T_1 \rightarrow T_2$

Returning to Sc

$$Sc=r_1(A)w_1(A)r_2(A)w_2(A)r_1(B)w_1(B)r_2(B)w_2(B)$$



$T_1 \rightarrow T_2$ $\qquad$ $T_1 \rightarrow T_2$

☛ no cycles ⇒ Sc is “equivalent” to a serial schedule (in this case T_1,T_2)

Concepts

Transaction: sequence of $r_i(x)$, $w_i(x)$ actions

Conflicting actions: $r_1(A)$ $w_2(A)$ $w_1(A)$
$w_2(A)$ $r_1(A)$ $w_2(A)$

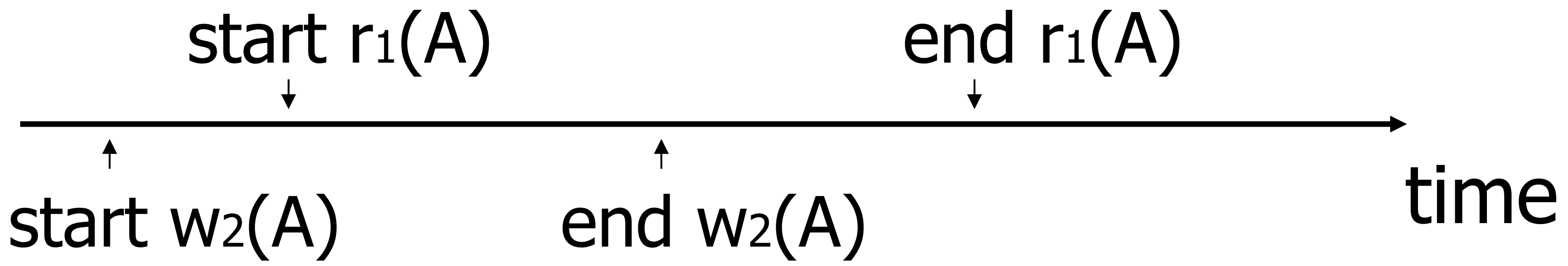
Schedule: represents chronological order in which actions are executed

Serial schedule: no interleaving of actions or transactions

Is it OK to model reads & writes as occurring at a single point in time in a schedule?

- S=… $r_1(x)$ … $w_2(b)$ …

What about conflicting, concurrent actions on same object?



start r_1(A)

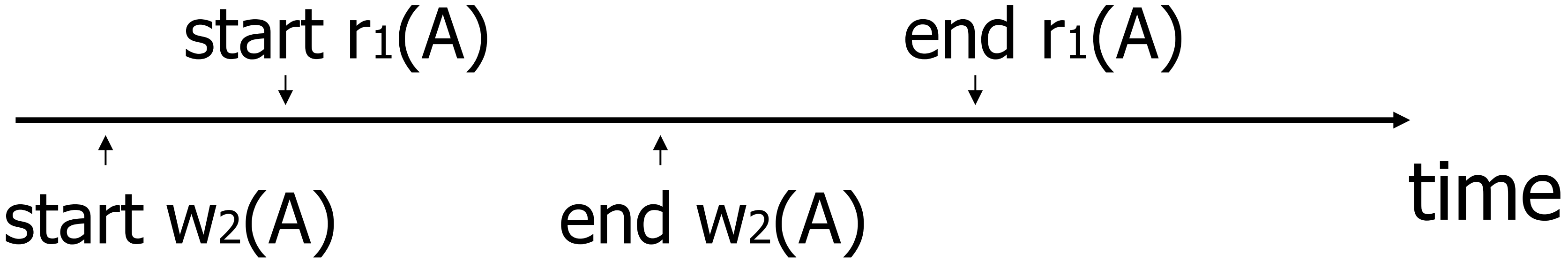
end r_1(A)

start w_2(A)

end w_2(A)

time

What about conflicting, concurrent actions on same object?



start $r_1(A)$

end $r_1(A)$

start $w_2(A)$

end $w_2(A)$

time

- Assume equivalent to either $r_1(A)$ $w_2(A)$
  or $w_2(A)$ $r_1(A)$
- ⇒ low level synchronization mechanism
- Assumption called "atomic actions"

Definition

S_1, S_2 are conflict equivalent schedules if S_1 can be transformed into S_2 by a series of swaps on non-conflicting actions.

Definition

A schedule is conflict serializable if it is conflict equivalent to some serial schedule.

Precedence graph P(S) (S is schedule)

Nodes: transactions in S

Arcs: Ti → Tj whenever

- $p_i(A)$, $q_j(A)$ are actions in S
- $p_i(A) <_S q_j(A)$
- at least one of p_i, q_j is a write

Exercise:

- What is P(S) for
  S = $w_3(A)\ w_2(C)\ r_1(A)\ w_1(B)\ r_1(C)\ w_2(A)\ r_4(A)\ w_4(D)$

- Is S serializable?

Another Exercise:

- What is P(S) for
  $S = w_1(A)\ r_2(A)\ \ r_3(A)\ w_4(A)$?

Lemma

S_1, S_2 conflict equivalent ⇒ $P(S_1)=P(S_2)$

Lemma

S_1, S_2 conflict equivalent $\Rightarrow P(S_1)=P(S_2)$

Proof:

Assume $P(S_1) \neq P(S_2)$

$\Rightarrow \exists\, T_i\!: T_i \rightarrow T_j$ in S_1 and not in S_2

$$\Rightarrow \begin{array}{l} S_1 = \ldots p_i(A) \ldots\ q_j(A) \ldots \\ \quad S_2 = \ldots q_j(A) \ldots p_i(A) \ldots \end{array} \qquad \left\{ \begin{array}{l} p_i,\, q_j \\ \text{conflict} \end{array} \right.$$

$\Rightarrow S_1, S_2$ not conflict equivalent

Note: $P(S_1)=P(S_2) \not\Rightarrow S_1, S_2$ conflict equivalent

Note: $P(S_1)=P(S_2) \not\Rightarrow S_1, S_2$ conflict equivalent

Counter example:

$S_1=w_1(A)\ r_2(A)\quad w_2(B)\ r_1(B)$

$S_2=r_2(A)\ w_1(A)\quad r_1(B)\ w_2(B)$

Theorem

$P(S_1)$ acyclic $\Longleftrightarrow$ S_1 conflict serializable

Theorem

$P(S_1)$ acyclic $\Longleftrightarrow$ S_1 conflict serializable

($\Leftarrow$) Assume S_1 is conflict serializable

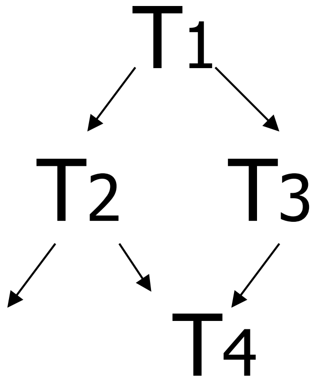
$\Rightarrow \exists\ S_s$: S_s, S_1 conflict equivalent

$\Rightarrow P(S_s) = P(S_1)$

$\Rightarrow P(S_1)$ acyclic since $P(S_s)$ is acyclic

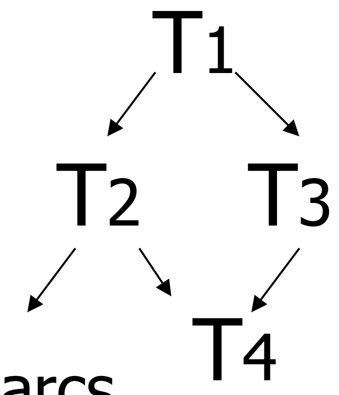
Theorem

$P(S_1)$ acyclic $\Longleftrightarrow$ S_1 conflict serializable



Theorem

$P(S_1)$ acyclic $\Longleftrightarrow$ S_1 conflict serializable



($\Rightarrow$) Assume $P(S_1)$ is acyclic

Transform S_1 as follows:

(1) Take T_1 to be transaction with no incident arcs

(2) Move all T_1 actions to the front

S_1 = ……. $q_j(A)$…….$p_1(A)$…..

(3) we now have S_1 = < T_1 actions ><... rest ...>

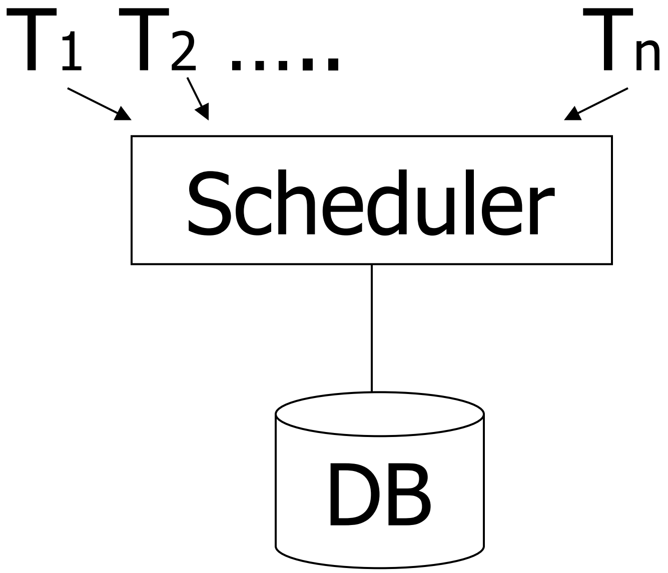
(4) repeat above steps to serialize rest!

How to enforce serializable schedules?

Option 1: run system, recording P(S);
at end of day, check for P(S)
cycles and declare if execution
was good

How to enforce serializable schedules?

Option 2: prevent P(S) cycles from occurring

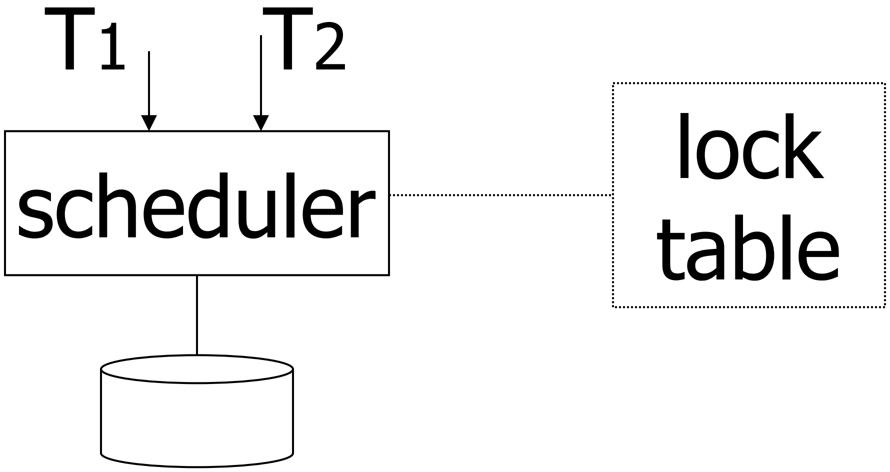


A locking protocol

Two new actions:

lock (exclusive): l_i (A)

unlock: u_i (A)



Rule #1: Well-formed transactions

T_i: … $l_i(A)$ … $p_i(A)$ … $u_i(A)$ ...

Rule #2 Legal scheduler

$$S = \ldots\ldots\ldots\; l_i(A)\; \underset{\text{no } l_j(A)}{\underleftrightarrow{\ldots\ldots\ldots\ldots\ldots}}\; u_i(A)\; \ldots\ldots\ldots$$

Exercise:

- What schedules are legal?
  What transactions are well-formed?

$S1 = l_1(A)l_1(B)r_1(A)w_1(B)l_2(B)u_1(A)u_1(B)$
$r_2(B)w_2(B)u_2(B)l_3(B)r_3(B)u_3(B)$

$S2 = l_1(A)r_1(A)w_1(B)u_1(A)u_1(B)$
$l_2(B)r_2(B)w_2(B)l_3(B)r_3(B)u_3(B)$

$S3 = l_1(A)r_1(A)u_1(A)l_1(B)w_1(B)u_1(B)$
$l_2(B)r_2(B)w_2(B)u_2(B)l_3(B)r_3(B)u_3(B)$

Exercise:

- What schedules are legal?
  What transactions are well-formed?

$S1 = l_1(A)l_1(B)r_1(A)w_1(B)l_2(B)u_1(A)u_1(B)$
$r_2(B)w_2(B)u_2(B)l_3(B)r_3(B)u_3(B)$

$S2 = l_1(A)r_1(A)w_1(B)u_1(A)u_1(B)$
$l_2(B)r_2(B)w_2(B)l_3(B)r_3(B)u_3(B)$ $u_2(B)$?

$S3 = l_1(A)r_1(A)u_1(A)l_1(B)w_1(B)u_1(B)$
$l_2(B)r_2(B)w_2(B)u_2(B)l_3(B)r_3(B)u_3(B)$

Schedule F

T1	T2
l_1(A);Read(A)	
A←A+100;Write(A);u_1(A)	
	l_2(A);Read(A)
	A←Ax2;Write(A);u_2(A)
	l_2(B);Read(B)
	B←Bx2;Write(B);u_2(B)
l_1(B);Read(B)	
B←B+100;Write(B);u_1(B)	

Schedule F

T1	T2	A	B
		25	25
l_1(A);Read(A)			
A←A+100;Write(A);u_1(A)		125	
	l_2(A);Read(A)		
	A←Ax2;Write(A);u_2(A)	250	
	l_2(B);Read(B)		
	B←Bx2;Write(B);u_2(B)		50
l_1(B);Read(B)			
B←B+100;Write(B);u_1(B)			150
		250	150

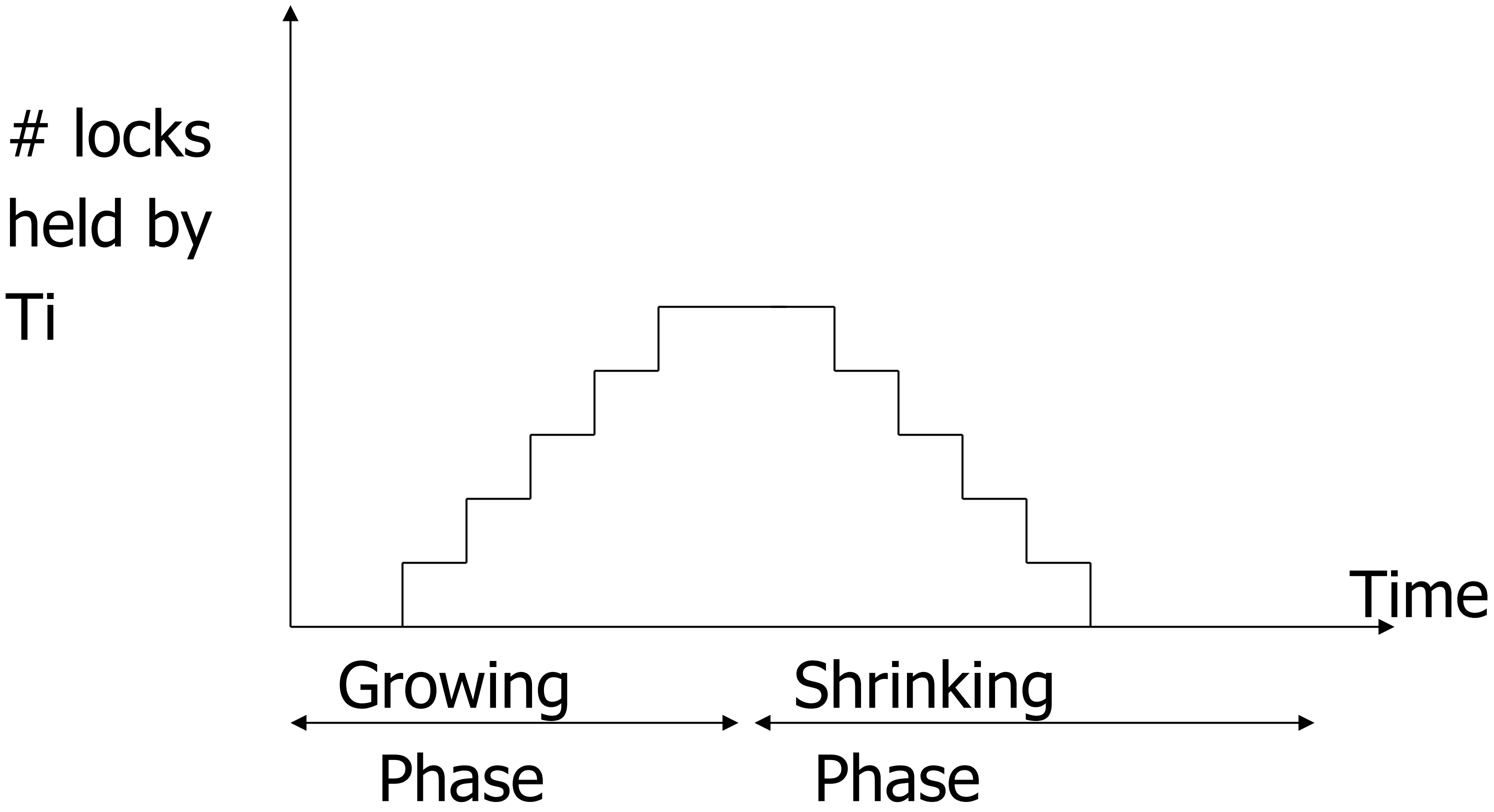
Rule #3 Two phase locking (2PL)

for transactions

T_i = ……. $l_i(A)$ ………… $u_i(A)$ ………

no unlocks

no locks


# locks
held by
Ti
Time
Growing
Phase
Shrinking
Phase

Schedule G

T1	T2
l_1(A);Read(A)	
A←A+100;Write(A)	
l_1(B); u_1(A)	
	l_2(A);Read(A)
	A←Ax2;Write(A);l_2(B)

delayed

Schedule G

T1	T2	
l_1(A);Read(A)		
A←A+100;Write(A)		
l_1(B); u_1(A)		
		delayed
	l_2(A);Read(A)	
	A←Ax2;Write(A); l_2(B)	
Read(B);B ← B+100		
Write(B); u_1(B)		

Schedule G

T1	T2
$l_1(A)$;Read(A)	
$A \leftarrow A+100$;Write(A)	
$l_1(B)$; $u_1(A)$	
	$l_2(A)$;Read(A)
	$A \leftarrow Ax2$;Write(A); $l_2(B)$ delayed
Read(B);$B \leftarrow B+100$	
Write(B); $u_1(B)$	
	$l_2(B)$; $u_2(A)$;Read(B)
	$B \leftarrow Bx2$;Write(B);$u_2(B)$;

Schedule H (T$_2$ reversed)

T1	T2
l_1(A); Read(A)	l_2(B);Read(B)
A ← A+100;Write(A)	B ← Bx2;Write(B)
l_1(B)	l_2(A)
delayed	delayed

- Assume deadlocked transactions are rolled back
  - They have no effect
  - They do not appear in schedule

E.g., Schedule H = $\underbrace{\qquad\qquad\qquad\qquad\qquad\qquad}_{\text{This space intentionally left blank!}}$

Next step:

Show that rules #1,2,3 ⇒ conflict-serializable schedules

Conflict rules for $l_i(A)$, $u_i(A)$:

- $l_i(A)$, $l_j(A)$ conflict
- $l_i(A)$, $u_j(A)$ conflict

Note: no conflict $< u_i(A), u_j(A)>$, $< l_i(A), r_j(A)>$,...

Theorem Rules #1,2,3 $\Rightarrow$ conflict serializable schedule (2PL)

Theorem Rules #1,2,3 ⇒ conflict
(2PL) serializable
schedule

To help in proof:

Definition Shrink(Ti) = SH(Ti) = first unlock action of Ti

Lemma

$Ti \rightarrow Tj$ in $S \Rightarrow SH(Ti) <_S SH(Tj)$

Lemma

$T_i \rightarrow T_j$ in S $\Rightarrow$ SH(Ti) $<_S$ SH(Tj)

Proof of lemma:

$T_i \rightarrow T_j$ means that

S = … p_i(A) … q_j(A) …; p,q conflict

By rules 1,2:

S = … p_i(A) … u_i(A) … l_j(A) ... q_j(A) …

Lemma

$T_i \rightarrow T_j$ in $S \Rightarrow SH(T_i) <_S SH(T_j)$

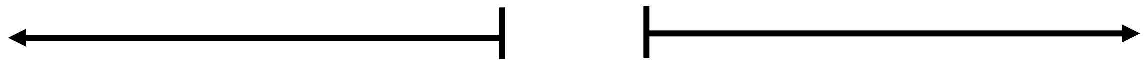
Proof of lemma:

$T_i \rightarrow T_j$ means that

$S = \ldots p_i(A) \ldots q_j(A) \ldots;$ p,q conflict

By rules 1,2:

$S = \ldots p_i(A) \ldots u_i(A) \ldots l_j(A) \ldots q_j(A) \ldots$



By rule 3: $SH(T_i)$ $SH(T_j)$

So, $SH(T_i) <_S SH(T_j)$

Theorem Rules #1,2,3 $\Rightarrow$ conflict
(2PL) serializable
schedule

Proof:

(1) Assume P(S) has cycle

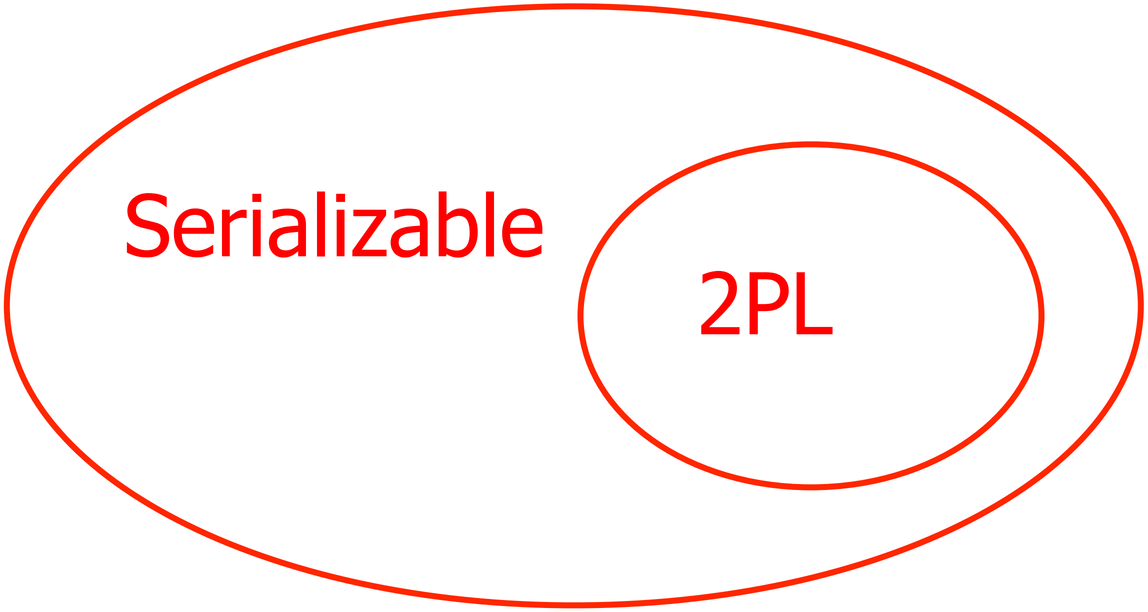
$$T_1 \rightarrow T_2 \rightarrow \ldots T_n \rightarrow T_1$$

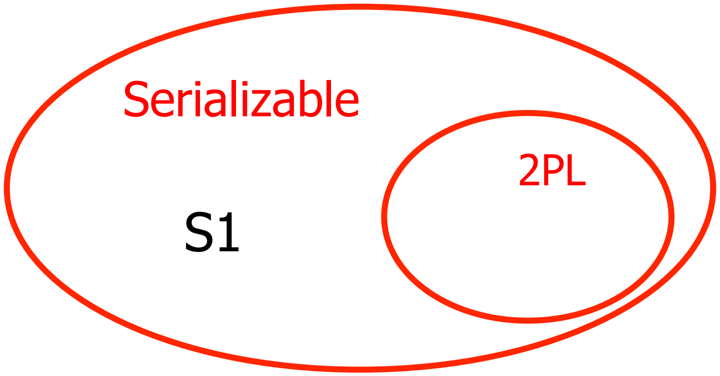
(2) By lemma: $SH(T_1) < SH(T_2) < \ldots < SH(T_1)$

(3) Impossible, so P(S) acyclic

(4) $\Rightarrow$ S is conflict serializable

2PL subset of Serializable





S1: w1(x)   w3(x)   w2(y)   w1(y)

S1: w1(x) w3(x) w2(y) w1(y)

- S1 cannot be achieved via 2PL:
  The lock by T1 for y must occur after w2(y), so the unlock by T1 for x must occur after this point (and before w1(x)). Thus, w3(x) cannot occur under 2PL where shown in S1 because T1 holds the x lock at that point.
- However, S1 is serializable
  (equivalent to T2, T1, T3).

If you need a bit more practice:

Are our schedules S_C and S_D 2PL schedules?

S_C: w1(A) w2(A) w1(B) w2(B)

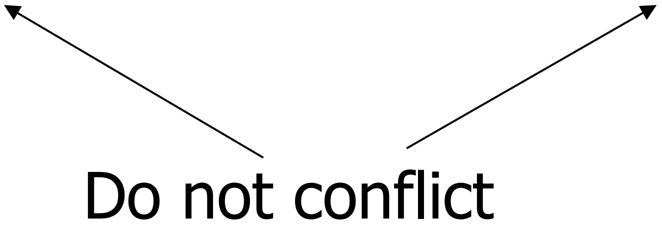
S_D: w1(A) w2(A) w2(B) w1(B)

- Beyond this simple 2PL protocol, it is all a matter of improving performance and allowing more concurrency….
  - Shared locks
  - Multiple granularity
  - Inserts, deletes and phantoms
  - Other types of C.C. mechanisms

Shared locks

So far:

S = ...$l_1(A)$ $r_1(A)$ $u_1(A)$ … $l_2(A)$ $r_2(A)$ $u_2(A)$ …



Do not conflict

Shared locks

So far:

$S = \ldots l_1(A)\ r_1(A)\ u_1(A)\ \ldots\ l_2(A)\ r_2(A)\ u_2(A)\ \ldots$

Do not conflict

Instead:

$S = \ldots\ ls_1(A)\ r_1(A)\ ls_2(A)\ r_2(A)\ \ldots.\ us_1(A)\ us_2(A)$

Lock actions

l-t_i(A): lock A in t mode (t is S or X)

u-t_i(A): unlock t mode (t is S or X)

Shorthand:

u_i(A): unlock whatever modes
T_i has locked A

Rule #1 Well formed transactions

T_i =... l-S_1(A) ... r_1(A) ... u_1 (A) ...

T_i =... l-X_1(A) ... w_1(A) ... u_1 (A) ...

- What about transactions that read and write same object?

Option 1: Request exclusive lock

T_i = ...l-X_1(A) … r_1(A) ... w_1(A) ... u(A) …

- What about transactions that read and write same object?

Option 2: Upgrade

(E.g., need to read, but don’t know if will write…)

T_i=... l-S_1(A) … r_1(A) ... l-X_1(A) …w_1(A) ...u(A)…

Think of
- Get 2nd lock on A, or
- Drop S, get X lock

Rule #2 Legal scheduler

S = ...$l\text{-}S_i(A)$ … … $u_i(A)$ …

$\longleftrightarrow$

no $l\text{-}X_j(A)$

S = ... $l\text{-}X_i(A)$ … … $u_i(A)$ …

$\longleftrightarrow$

no $l\text{-}X_j(A)$
no $l\text{-}S_j(A)$

A way to summarize Rule #2

Compatibility matrix

Comp

	S	X
S	true	false
X	false	false

Rule # 3 2PL transactions

No change except for upgrades:

(I) If upgrade gets more locks
(e.g., S → {S, X}) then no change!

(II) If upgrade releases read (shared)
lock (e.g., S → X)
- can be allowed in growing phase

Theorem Rules 1,2,3 ⇒ Conf.serializable
for S/X locks schedules

Proof: similar to X locks case

Detail:

l-t_i(A), l-r_j(A) do not conflict if comp(t,r)

l-t_i(A), u-r_j(A) do not conflict if comp(t,r)

Lock types beyond S/X

Examples:

(1) increment lock

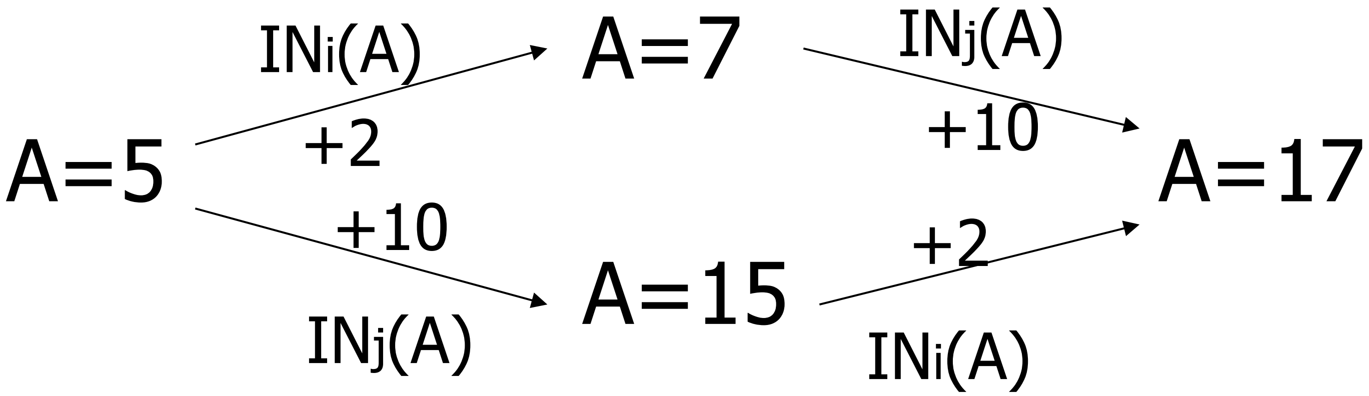
(2) update lock

Example (1): increment lock

- Atomic increment action: $IN_i(A)$

  {Read(A); A ← A+k; Write(A)}

- $IN_i(A)$, $IN_j(A)$ do not conflict!



Comp

	S	X	I
S			
X			
I			

Comp

	S	X	I
S	T	F	F
X	F	F	F
I	F	F	T

Update locks

A common deadlock problem with upgrades:

T1	T2
l-S_1(A)	
	l-S_2(A)
l-X_1(A)	
	l-X_2(A)

--- Deadlock ---

Solution

If T_i wants to read A and knows it may later want to write A, it requests update lock (not shared)

Comp

New request

Lock already held in

	S	X	U
S			
X			
U			

New request

Comp	S	X	U
S	T	F	T
X	F	F	F
U	TorF	F	F

Lock already held in

-> symmetric table?

Note: object A may be locked in different modes at the same time...

$$S_1=\ldots l\text{-}S_1(A)\ldots l\text{-}S_2(A)\ldots l\text{-}U_3(A)\ldots \left\{ \begin{array}{l} l\text{-}S_4(A)\ldots ? \\ l\text{-}U_4(A)\ldots ? \end{array}\right.$$

Note: object A may be locked in different modes at the same time...

$$S_1 = \ldots l\text{-}S_1(A) \ldots l\text{-}S_2(A) \ldots l\text{-}U_3(A) \ldots \begin{cases} l\text{-}S_4(A) \ldots ? \\ l\text{-}U_4(A) \ldots ? \end{cases}$$

- To grant a lock in mode t, mode t must be compatible with all currently held locks on object

How does locking work in practice?

- Every system is different

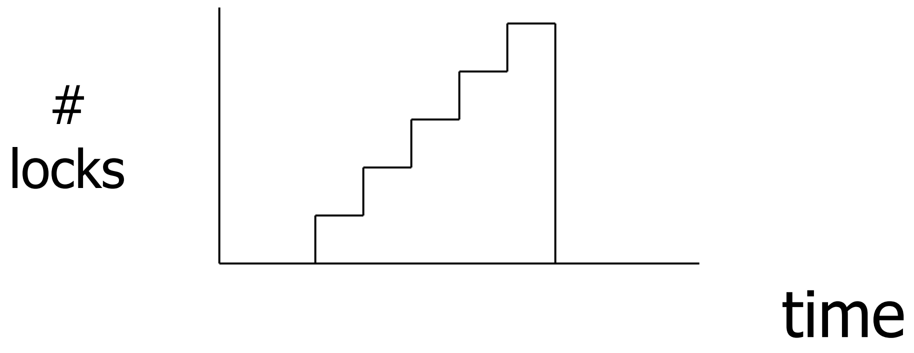
  (E.g., may not even provide
  CONFLICT-SERIALIZABLE schedules)

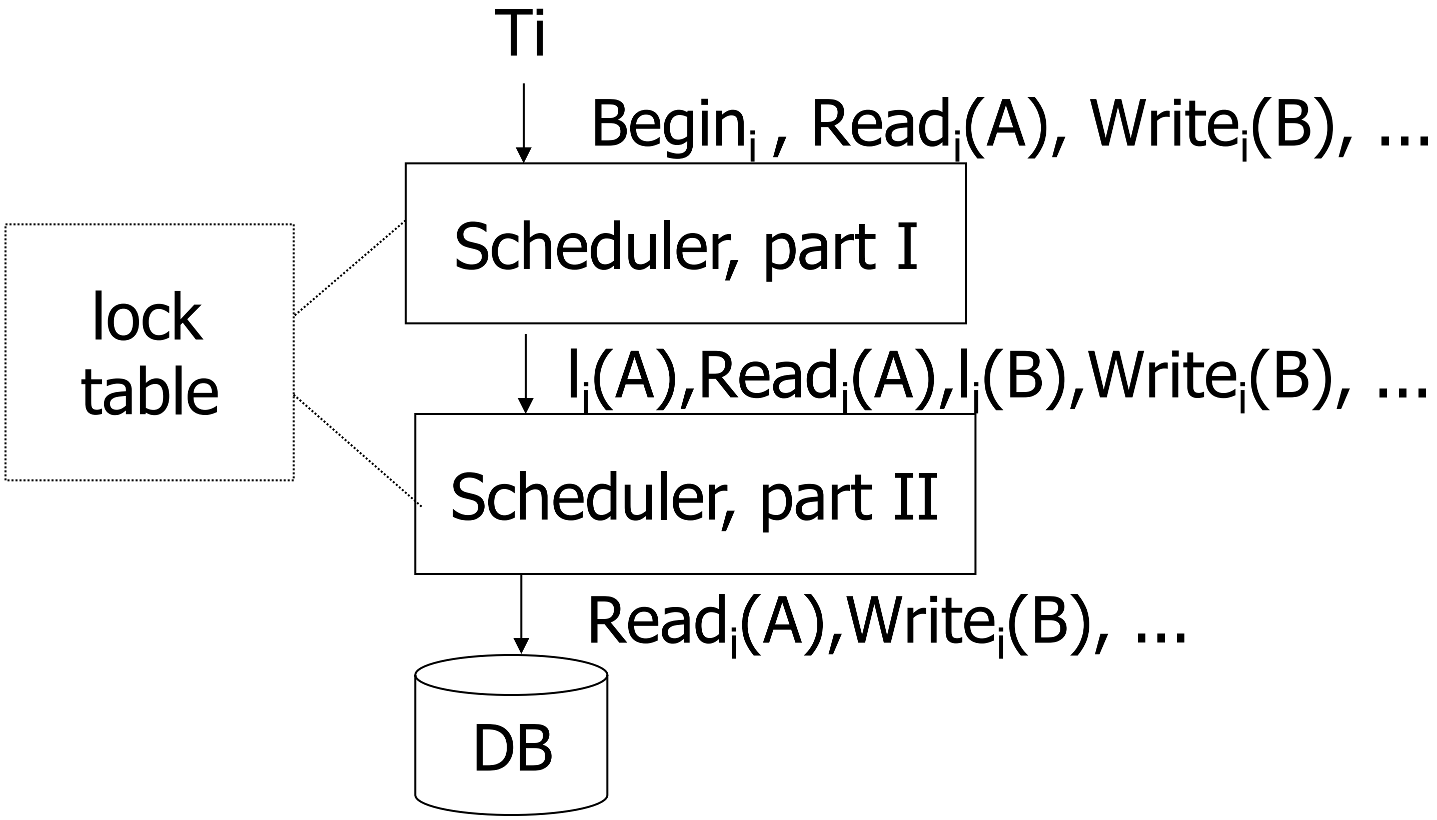
- But here is one (simplified) way ...

Sample Locking System:

(1) Don't trust transactions to request/release locks

(2) Hold all locks until transaction commits





Ti

Begin_i , $\text{Read}_i(A)$, $\text{Write}_i(B)$, ...

lock table

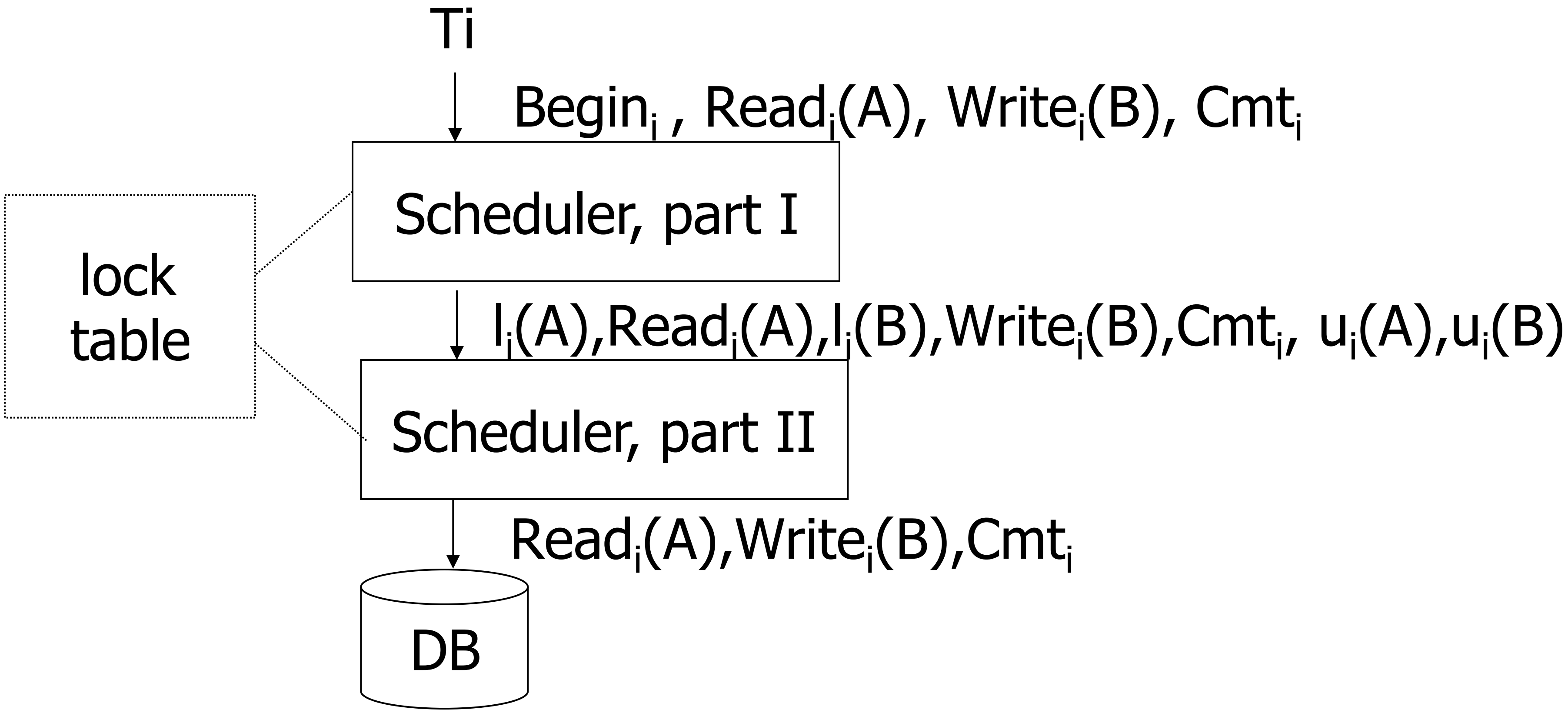
Scheduler, part I

$l_i(A)$,$\text{Read}_i(A)$,$l_i(B)$,$\text{Write}_i(B)$, ...

Scheduler, part II

$\text{Read}_i(A)$,$\text{Write}_i(B)$, ...

DB



Ti

Begin_i , Read_i(A), Write_i(B), Cmt_i

Scheduler, part I

lock table

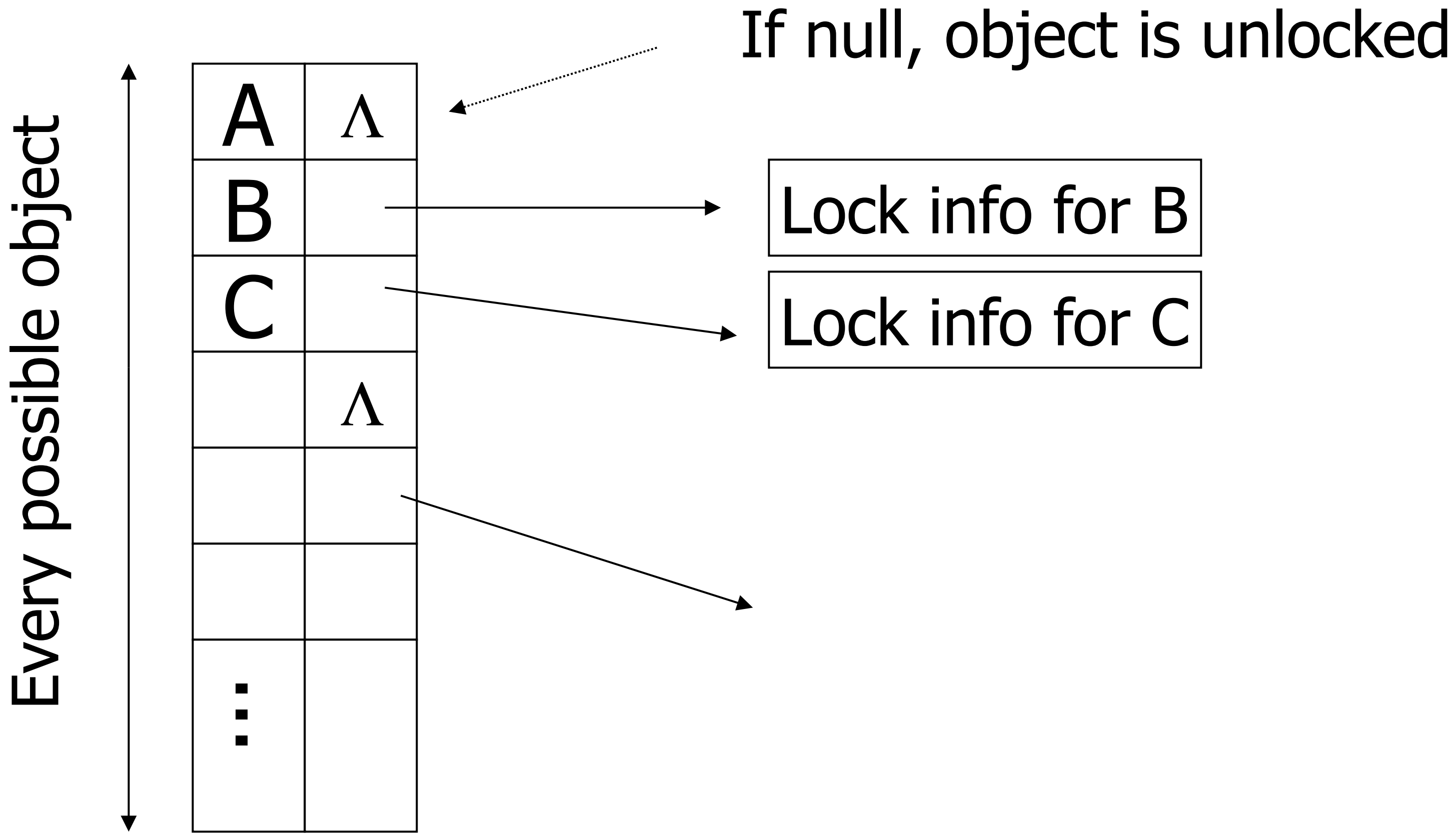
l_i(A),Read_i(A),l_i(B),Write_i(B),Cmt_i, u_i(A),u_i(B)

Scheduler, part II

Read_i(A),Write_i(B),Cmt_i

DB

Lock table   Conceptually



If null, object is unlocked

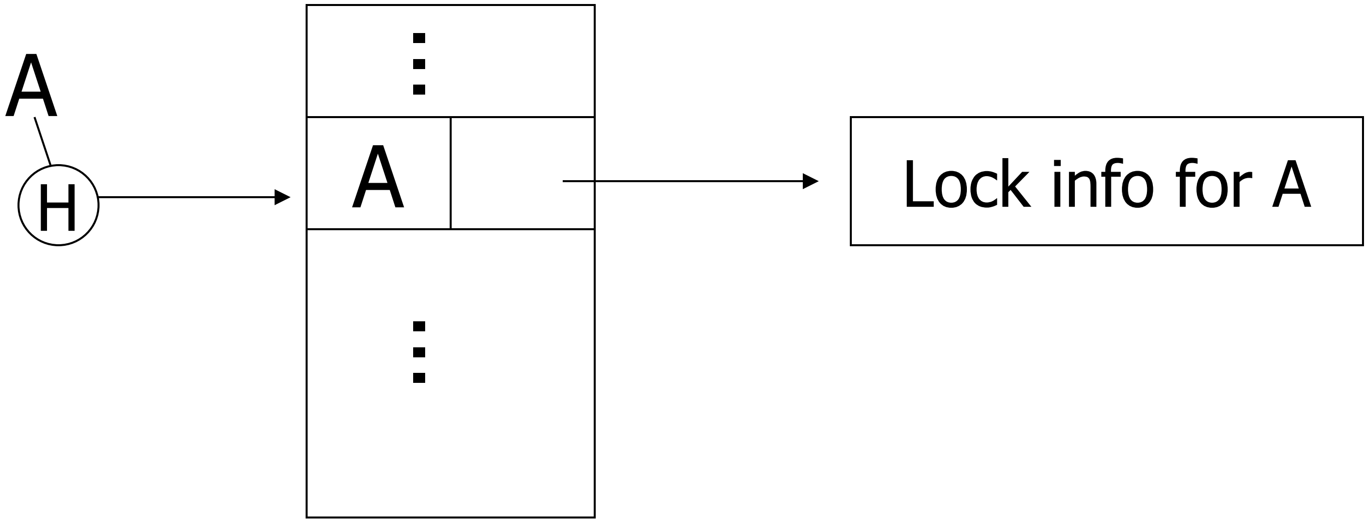
Every possible object

A	Λ
B	
C	
	Λ
⋮	

Lock info for B

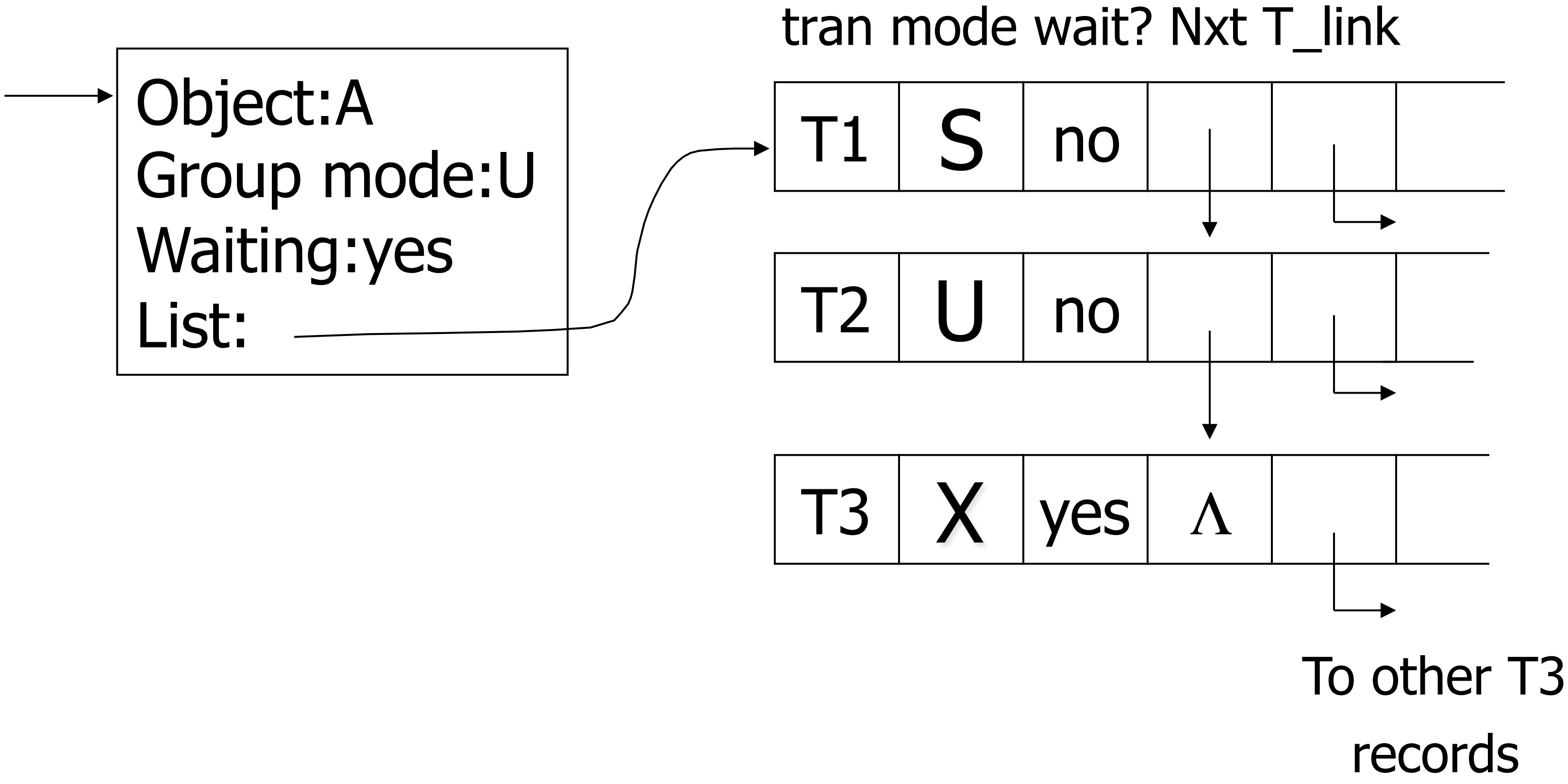
Lock info for C

But use hash table:



If object not found in hash table, it is unlocked

Lock info for A - example

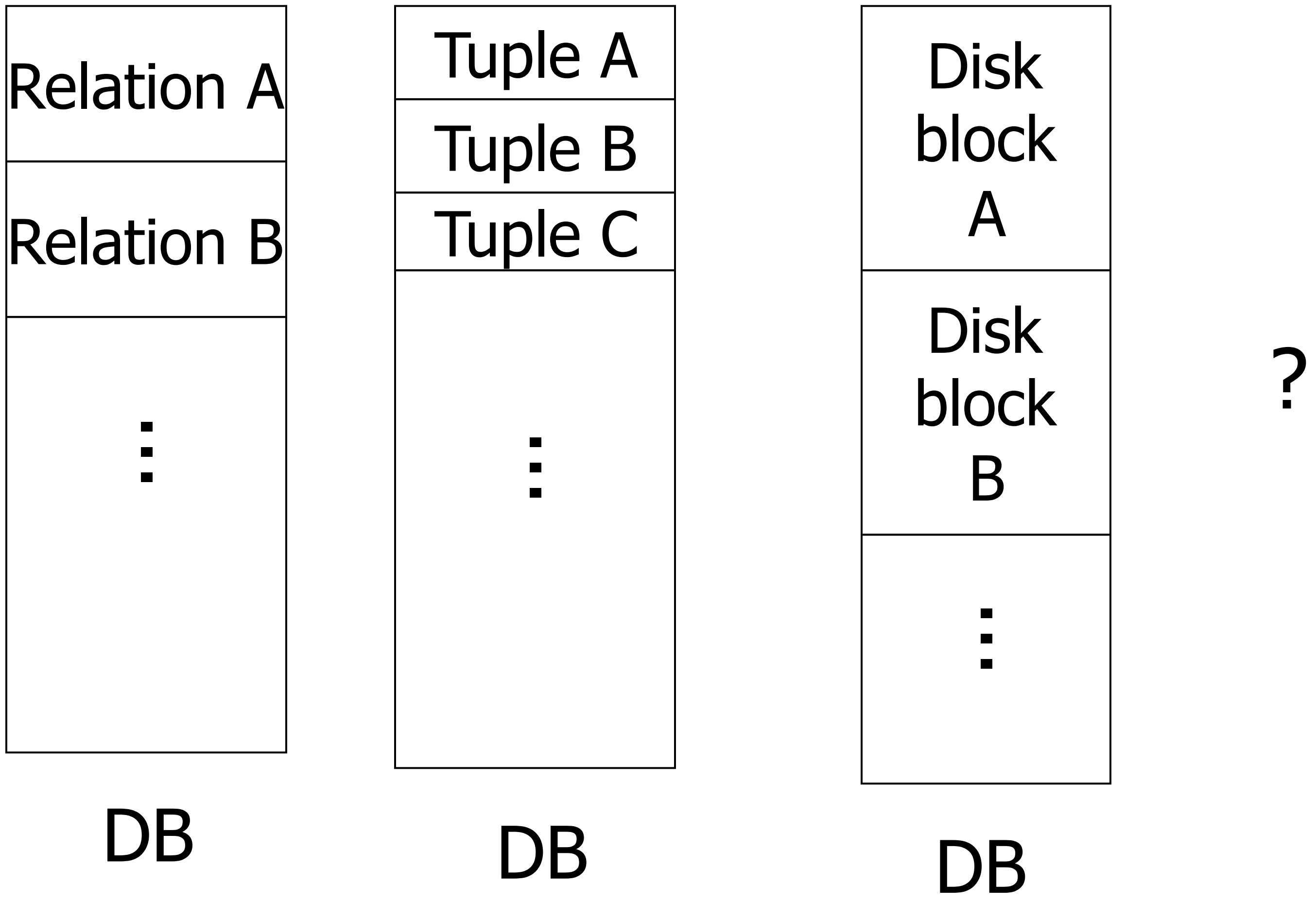


Object:A
Group mode:U
Waiting:yes
List:

tran	mode	wait?	Nxt	T_link
T1	S	no		
T2	U	no		
T3	X	yes	Λ	

To other T3 records

What are the objects we lock?



Relation A

Relation B

⋮

DB

Tuple A

Tuple B

Tuple C

⋮

DB

Disk block A

Disk block B

⋮

DB

?

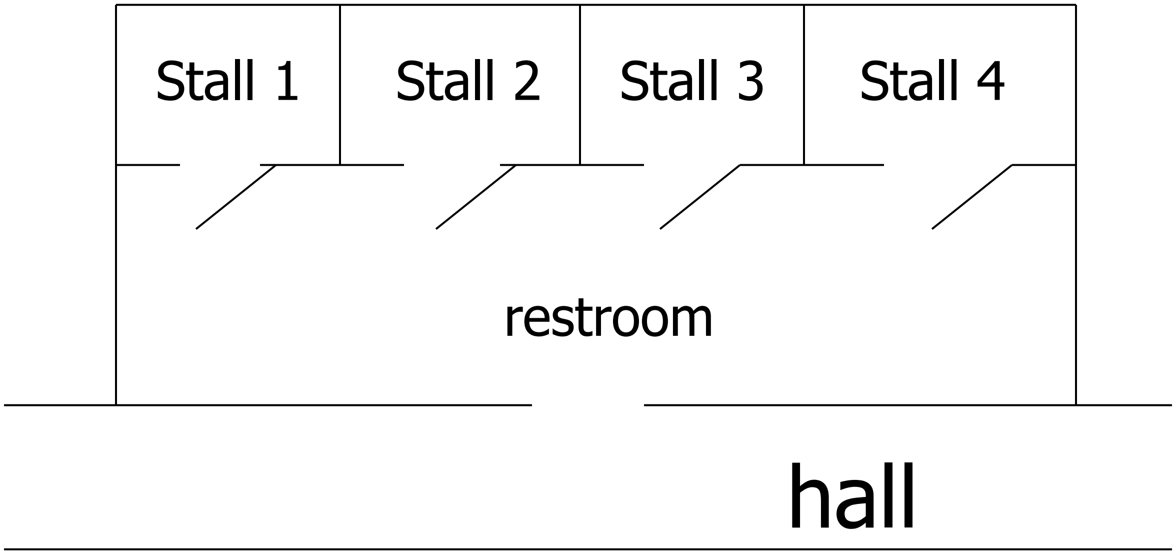
- Locking works in any case, but should we choose small or large objects?

- Locking works in any case, but should we choose small or large objects?

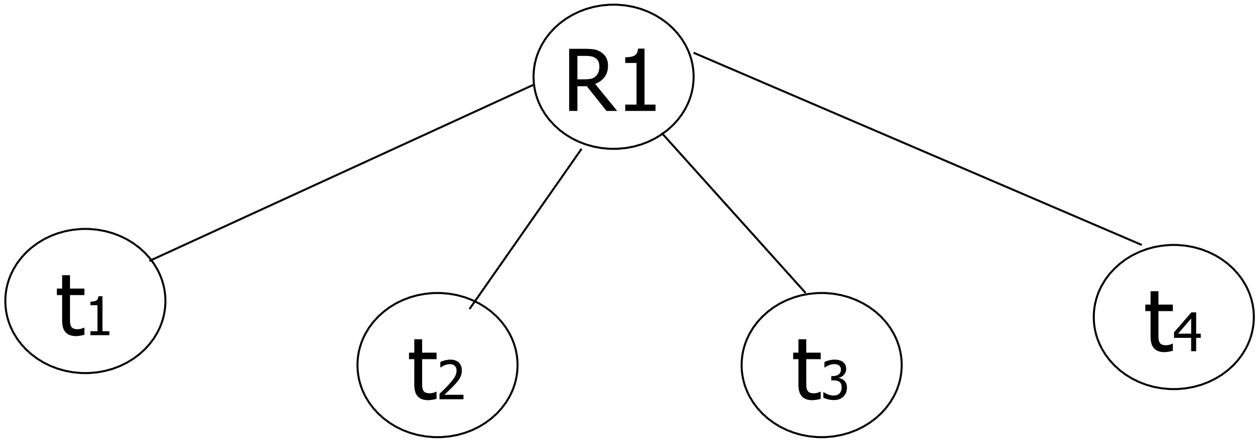
- If we lock large objects (e.g., Relations)
  - Need few locks
  - Low concurrency
- If we lock small objects (e.g., tuples,fields)
  - Need more locks
  - More concurrency

We <u>can</u> have it both ways!!

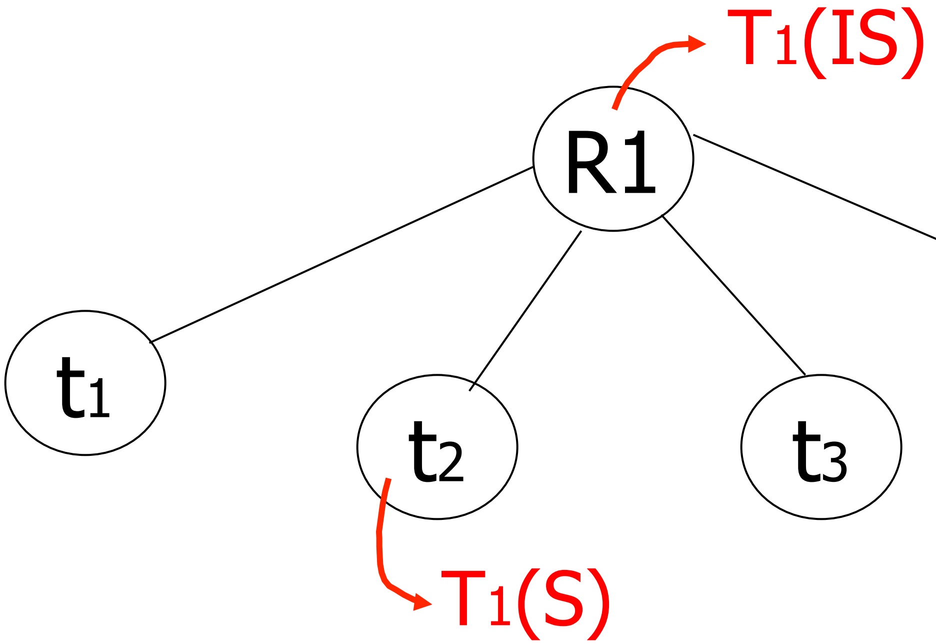
Ask any janitor to give you the solution...



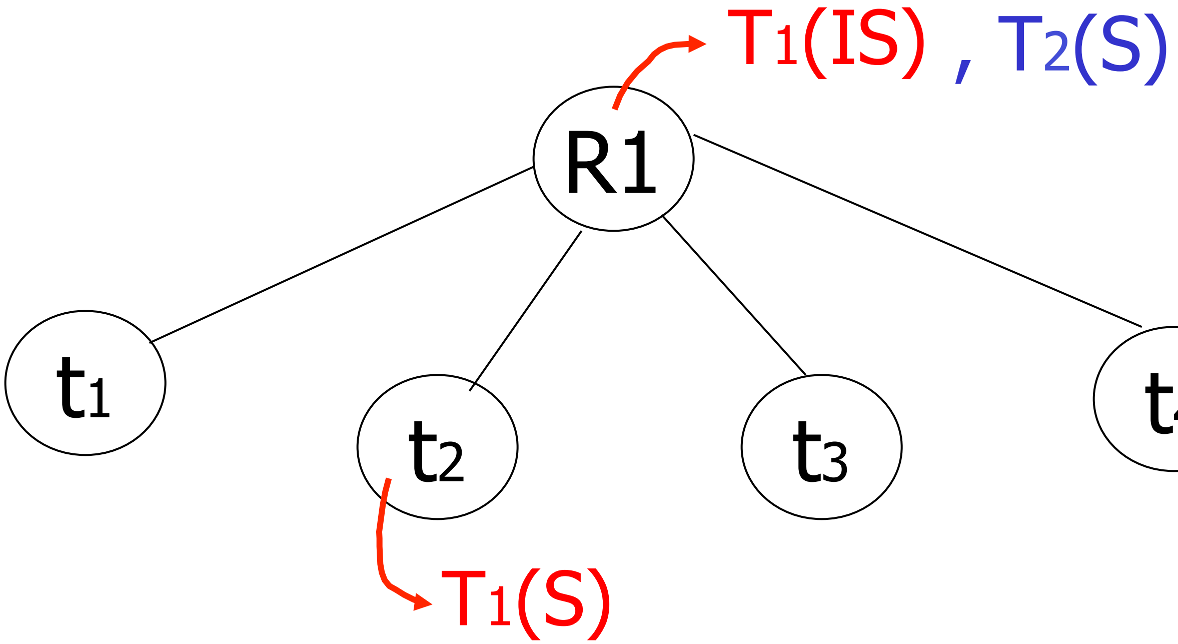
Example



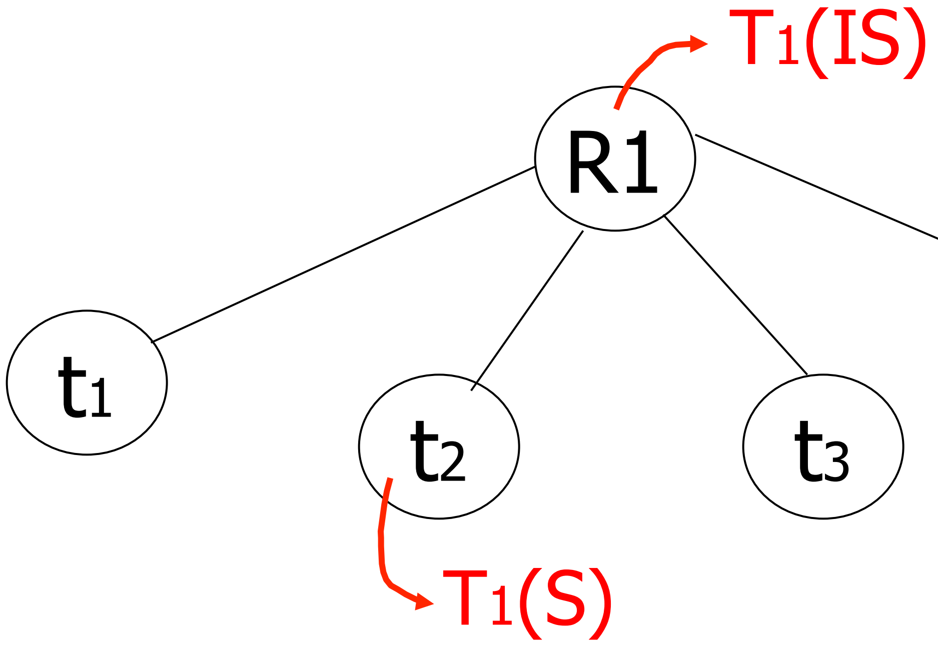
Example



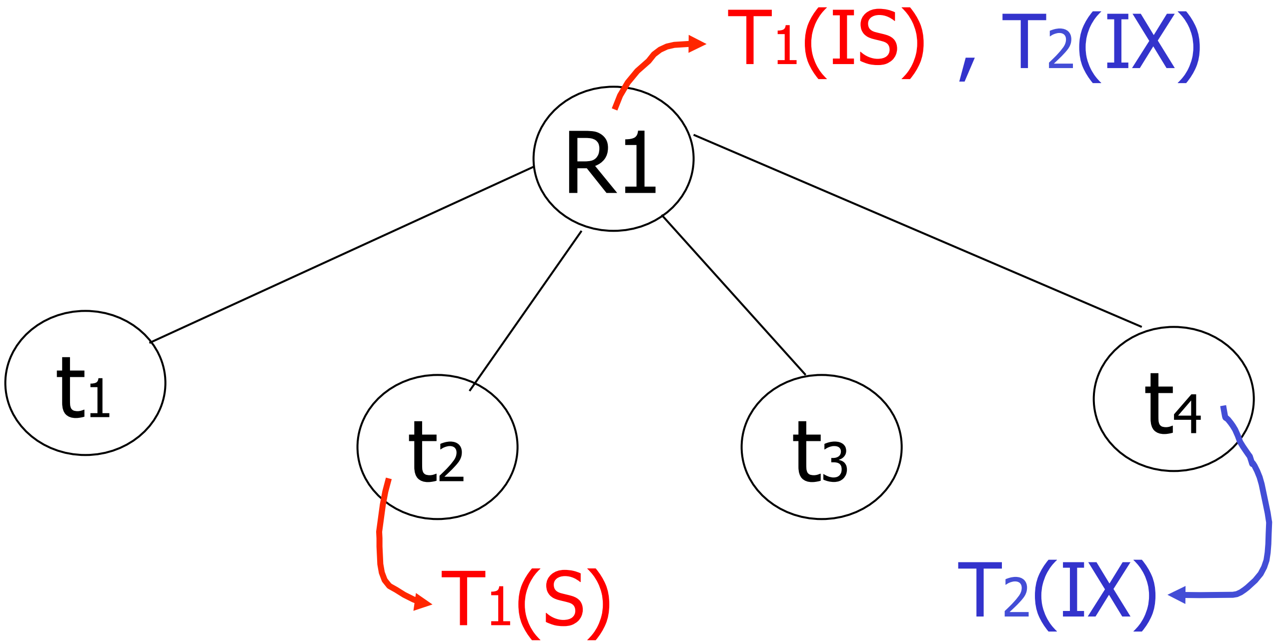
Example



Example (b)



Example



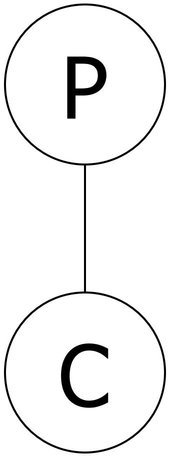
Multiple granularity

Comp		Requestor				
		IS	IX	S	SIX	X
	IS					
Holder	IX					
	S					
	SIX					
	X					

Multiple granularity

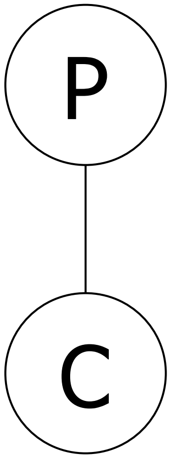
Comp		Requestor				
		IS	IX	S	SIX	X
	IS	T	T	T	T	F
Holder	IX	T	T	F	F	F
	S	T	F	T	F	F
	SIX	T	F	F	F	F
	X	F	F	F	F	F

Parent locked in	Child can be locked in
IS	
IX	
S	
SIX	
X	



Parent locked in	Child can be locked by same transaction in
IS	IS, S
IX	IS, S, IX, X, SIX
S	none
SIX	X, IX, [SIX]
X	none

not necessary



Rules

(1) Follow multiple granularity comp function

(2) Lock root of tree first, any mode

(3) Node Q can be locked by Ti in S or IS only if parent(Q) locked by Ti in IX or IS

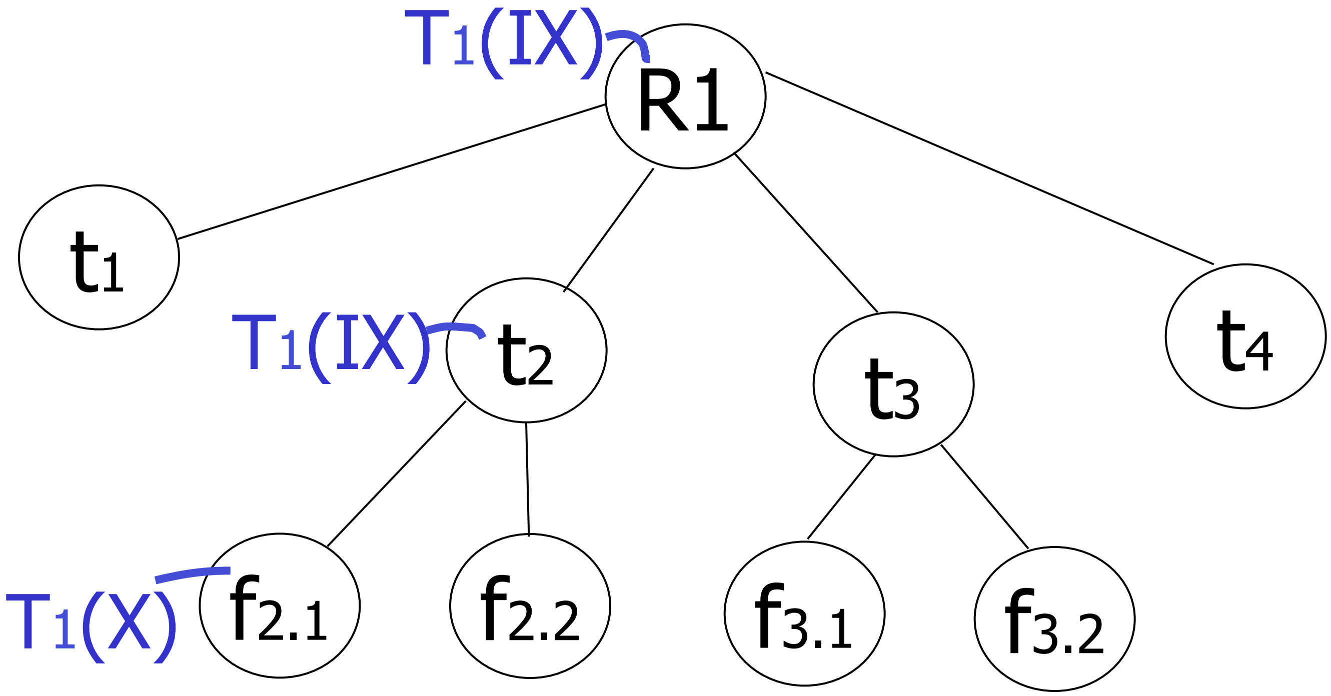
(4) Node Q can be locked by Ti in X,SIX,IX only if parent(Q) locked by Ti in IX,SIX

(5) Ti is two-phase

(6) Ti can unlock node Q only if none of Q's children are locked by Ti

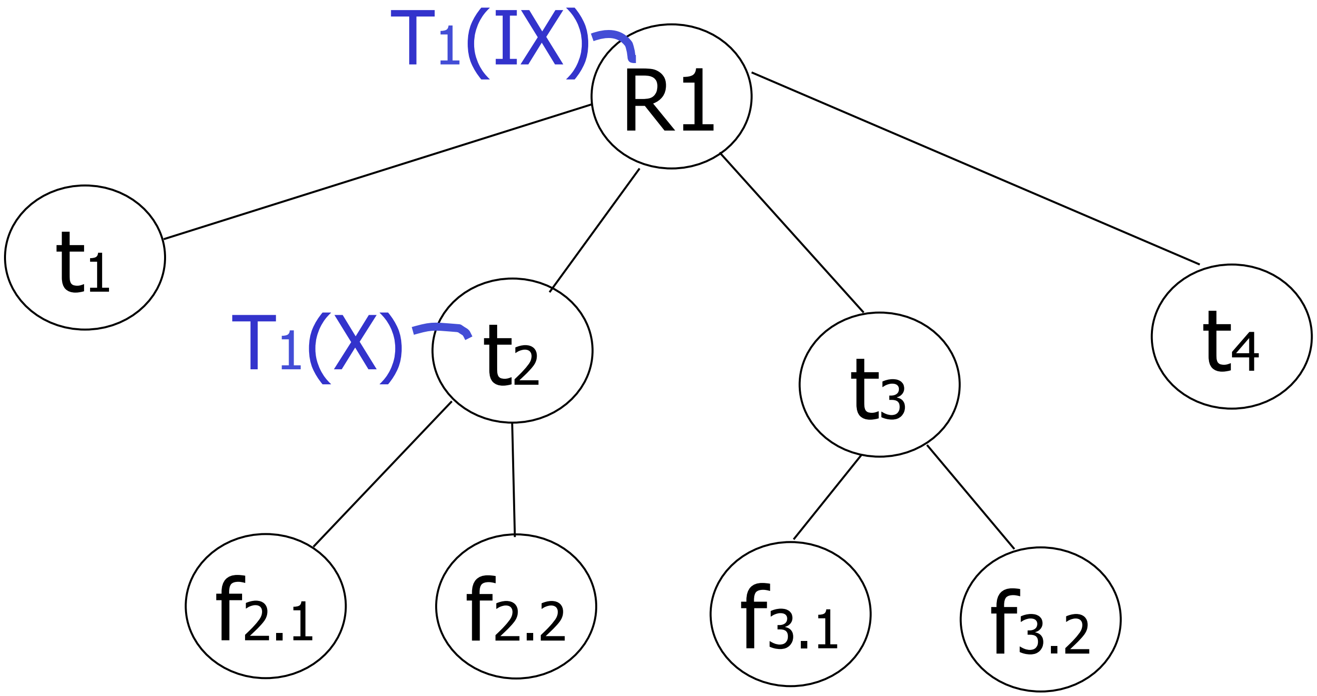
Exercise:

- Can T_2 access object $f_{2.2}$ in X mode? What locks will T_2 get?



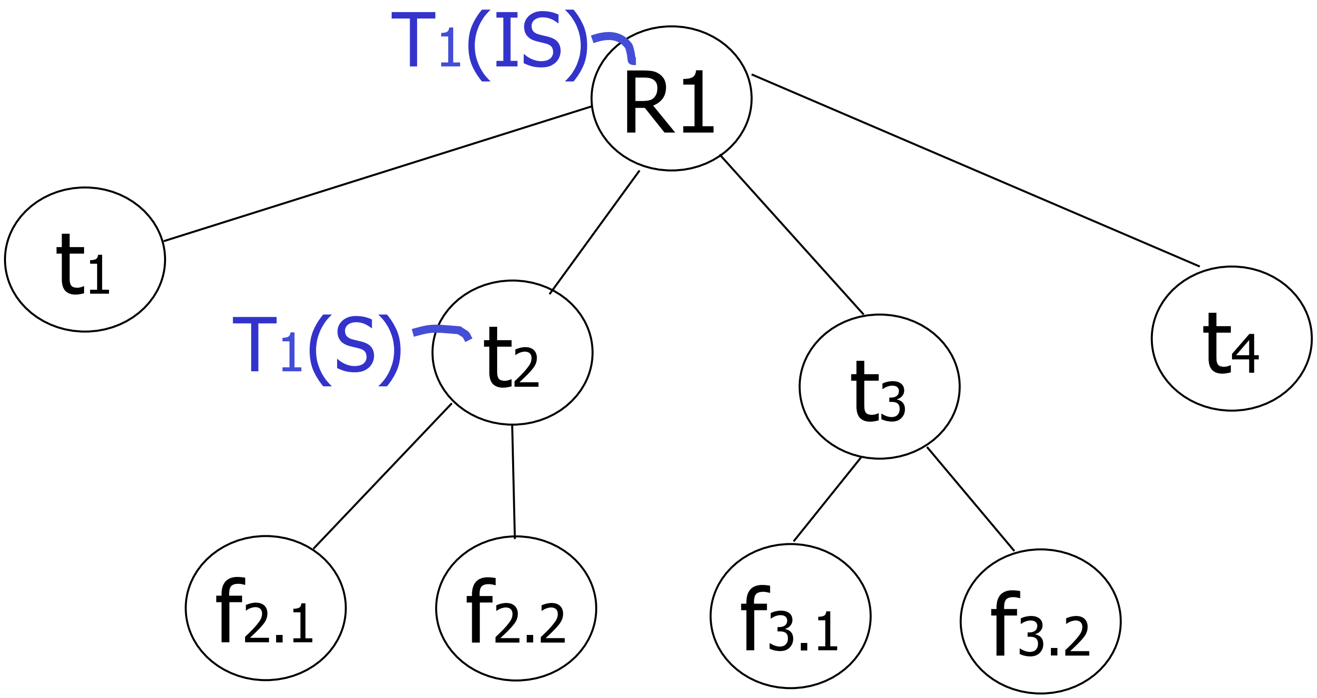
Exercise:

- Can T_2 access object $f_{2.2}$ in X mode? What locks will T_2 get?



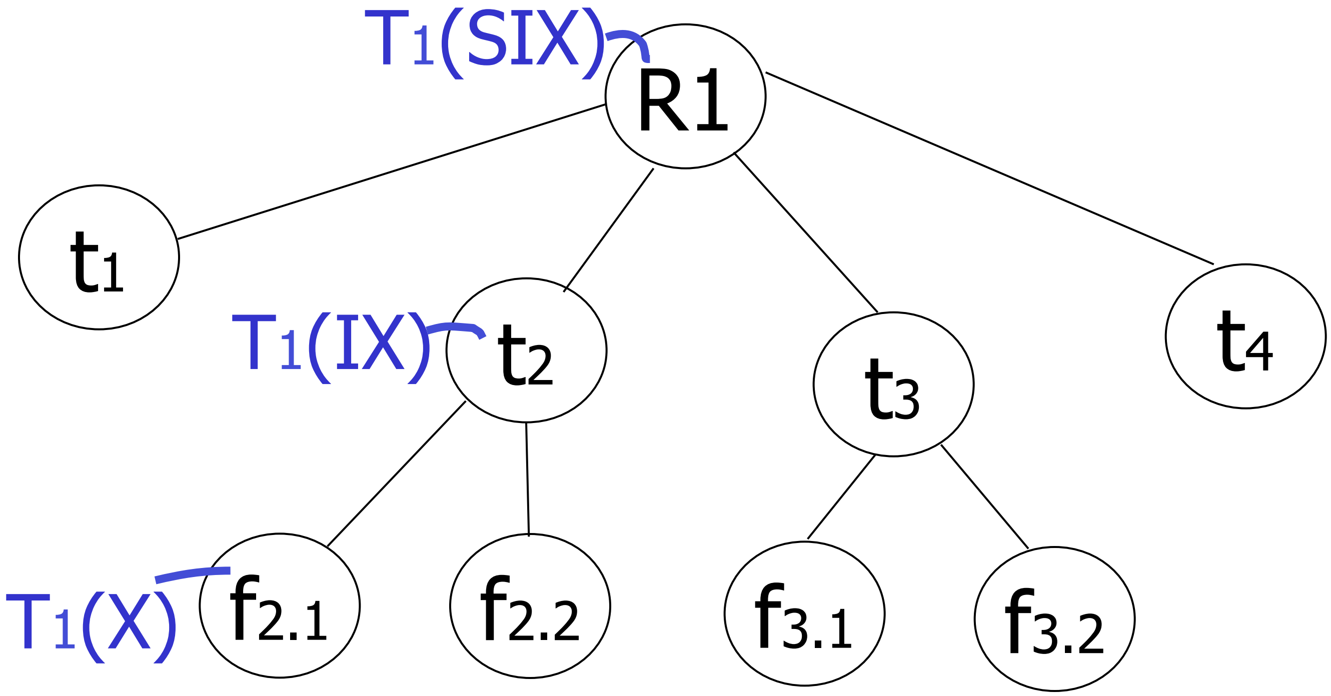
Exercise:

- Can T_2 access object $f_{3.1}$ in X mode? What locks will T_2 get?



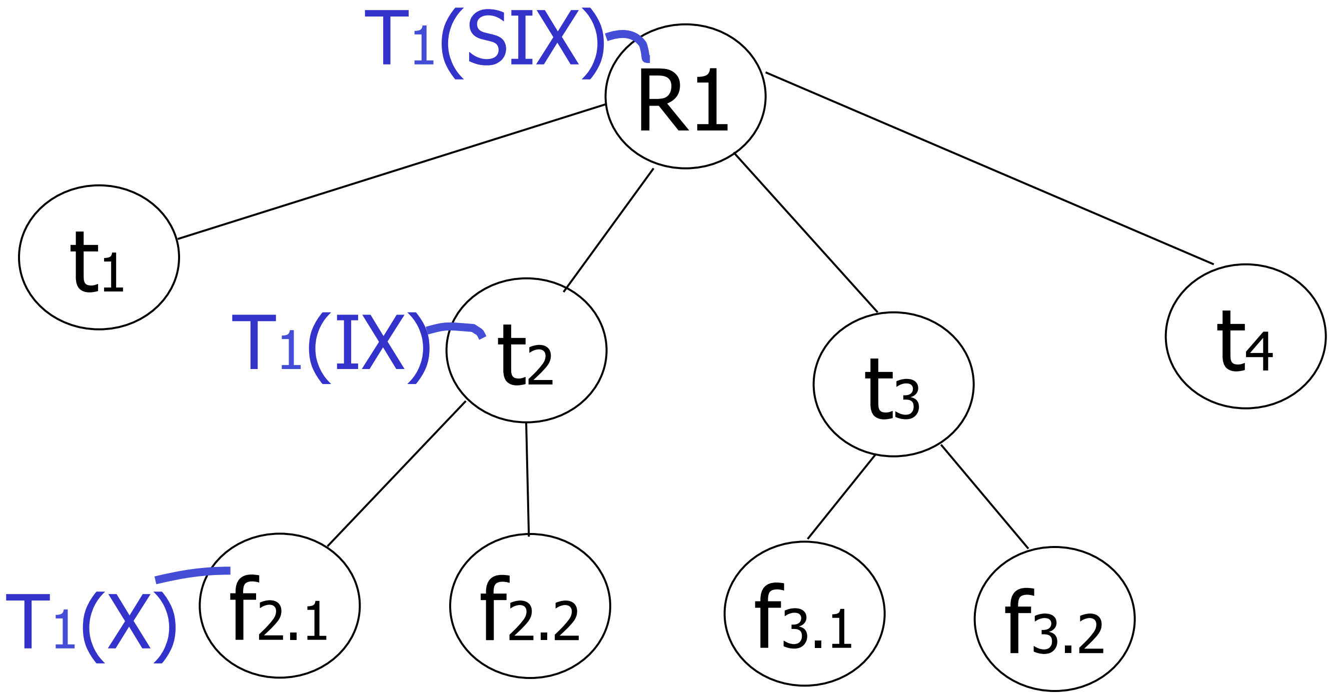
Exercise:

- Can T_2 access object $f_{2.2}$ in S mode? What locks will T_2 get?

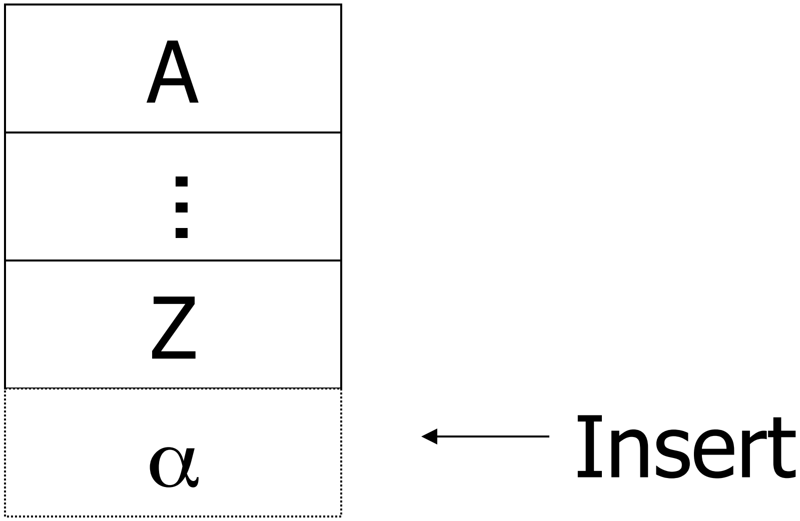


Exercise:

- Can T_2 access object $f_{2.2}$ in X mode? What locks will T_2 get?



Insert + delete operations



Modifications to locking rules:

(1) Get exclusive lock on A before deleting A

(2) At insert A operation by Ti, Ti is given exclusive lock on A

Still have a problem: **Phantoms**

Example: relation R (E#,name,…)
constraint: E# is key
use tuple locking

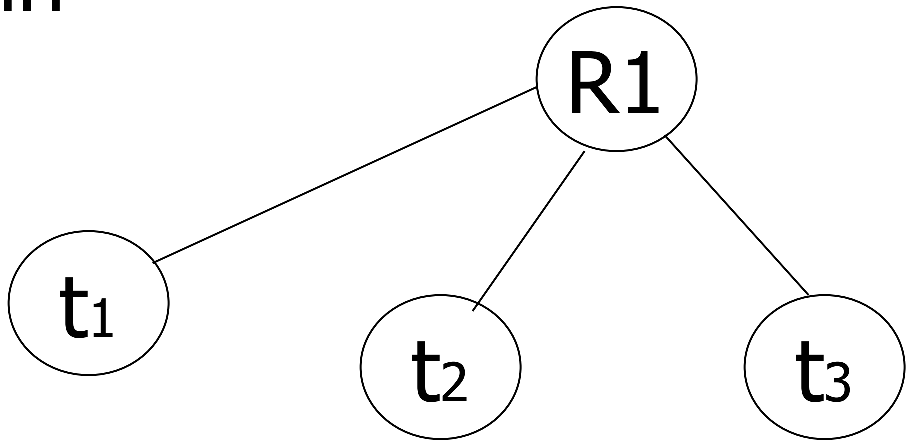
R	E#	Name	….
o1	55	Smith	
o2	75	Jones	

T_1: Insert <12,Obama,…> into R
T_2: Insert <12,Romney,…> into R

T_1	T_2
$S_1(o_1)$	$S_2(o_1)$
$S_1(o_2)$	$S_2(o_2)$
Check Constraint	Check Constraint
⋮	⋮
Insert o_3[12,Obama,..]	
	Insert o_4[12,Romney,..]

Solution

- Use multiple granularity tree
- Before insert of node Q,
  lock parent(Q) in
  X mode



Back to example

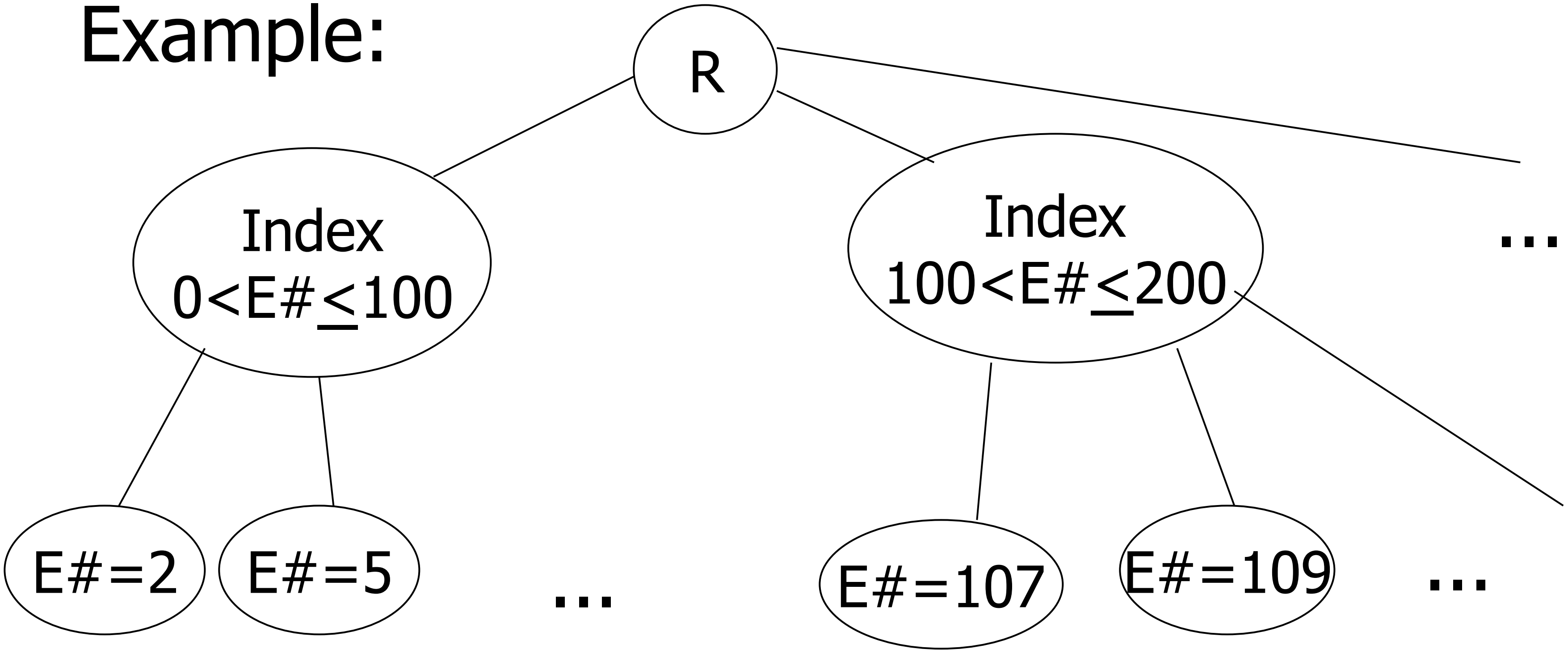
T1: Insert<12,Obama>

T2: Insert<12,Romney>

T_1	T_2
$X_1(R)$	
	$X_2(R)$ *delayed*
Check constraint	
Insert<12,Obama>	
$U_1(R)$	
	$X_2(R)$
	Check constraint
	Oops! e# = 12 already in R!

Instead of using R, can use index on R:

Example:



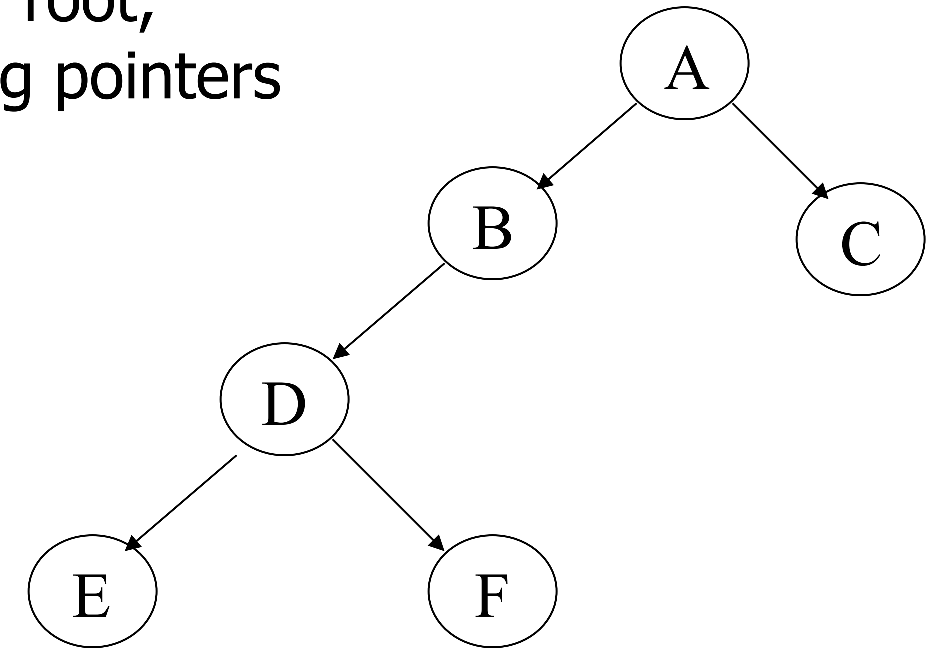
- This approach can be generalized to multiple indexes...

Next:

- Tree-based concurrency control
- Validation concurrency control

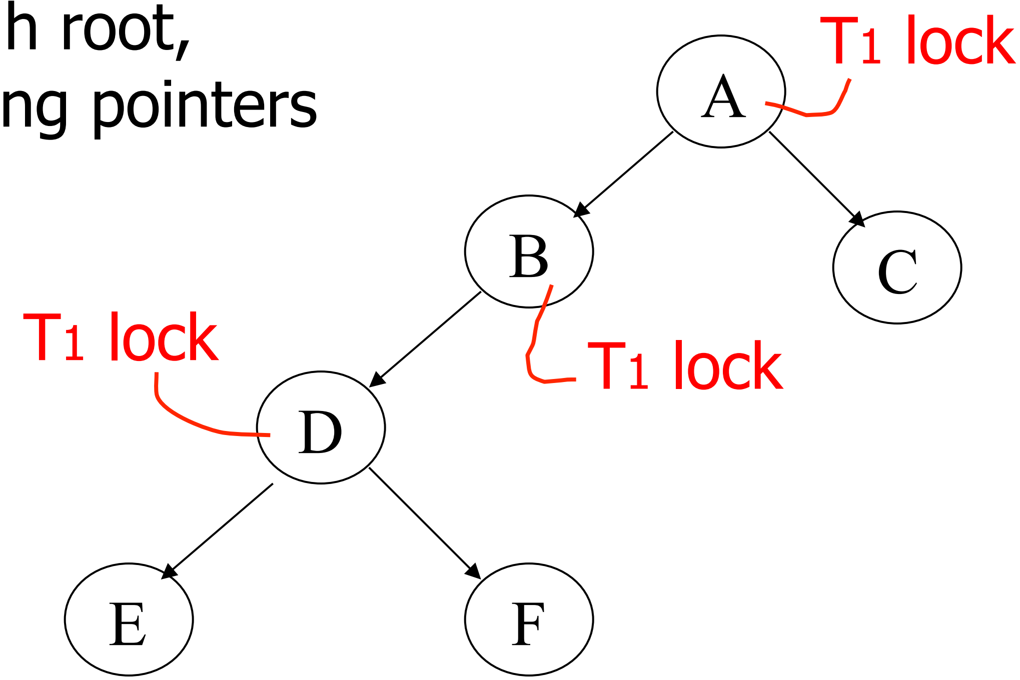
Example

- all objects accessed through root, following pointers



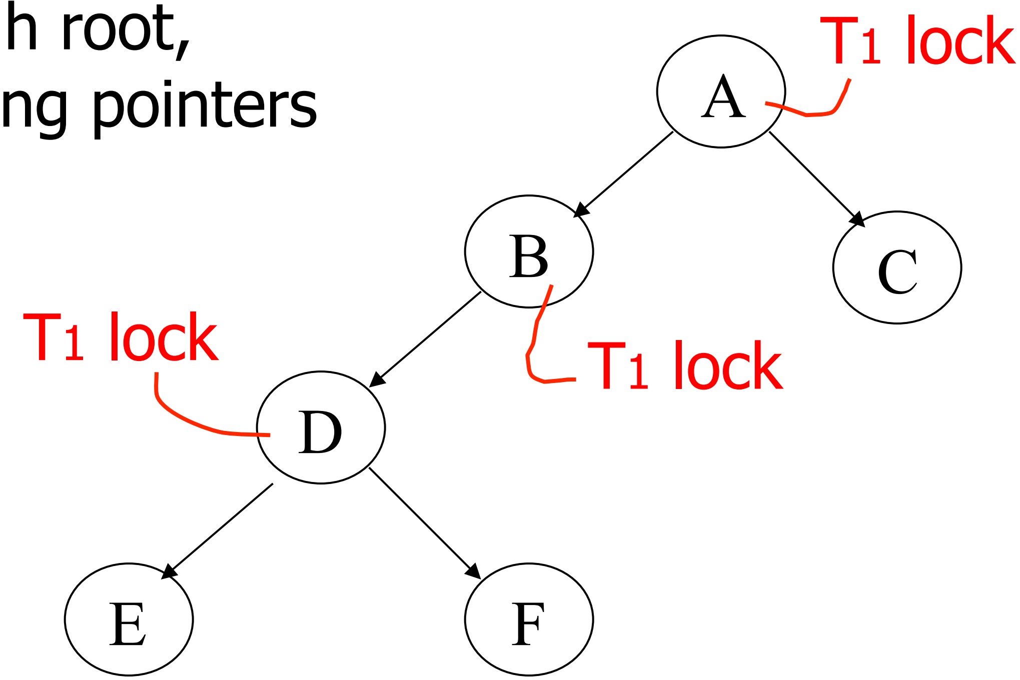
Example

- all objects accessed through root, following pointers



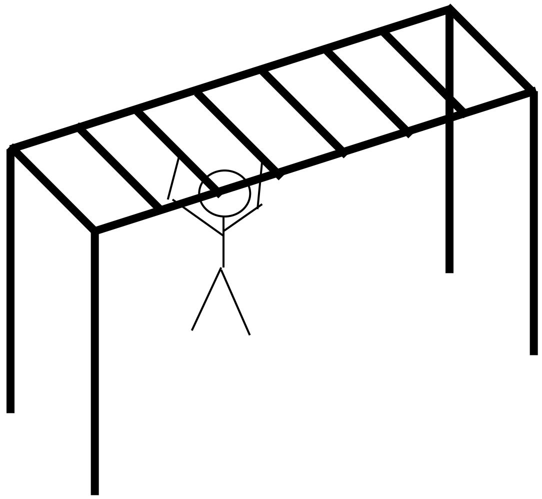
Example

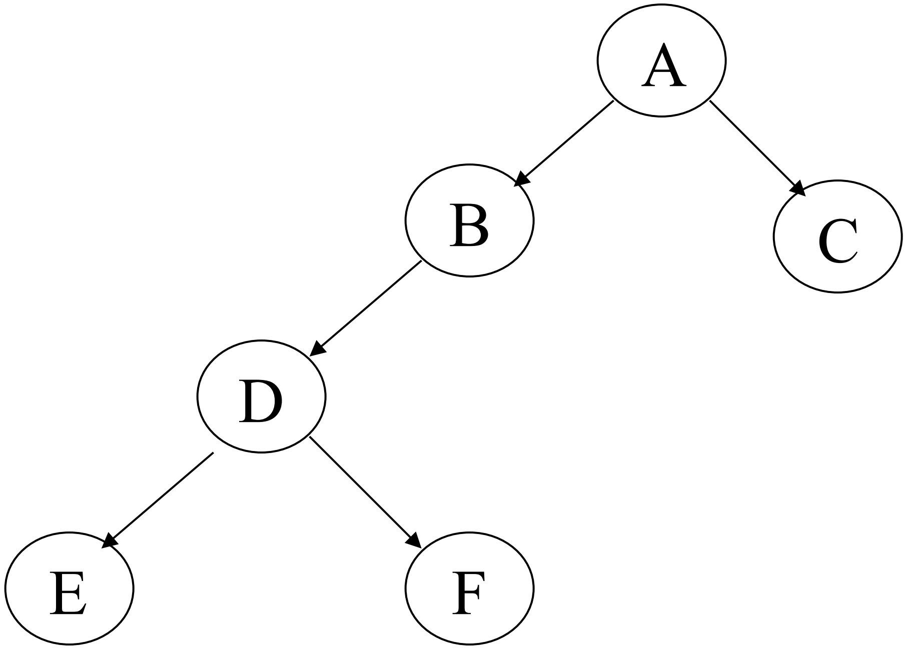
- all objects accessed through root, following pointers



☛ can we release A lock if we no longer need A??

Idea: traverse like "Monkey Bars"





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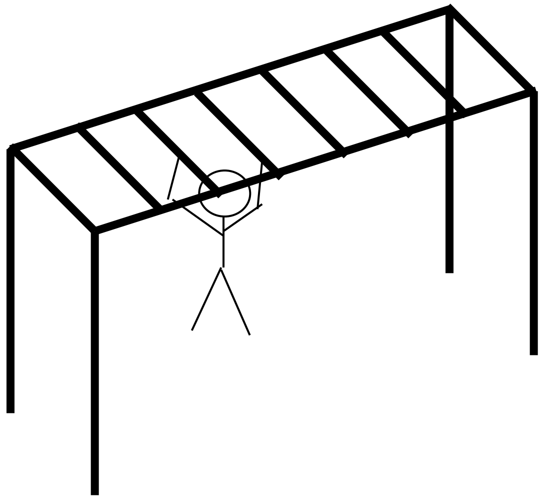

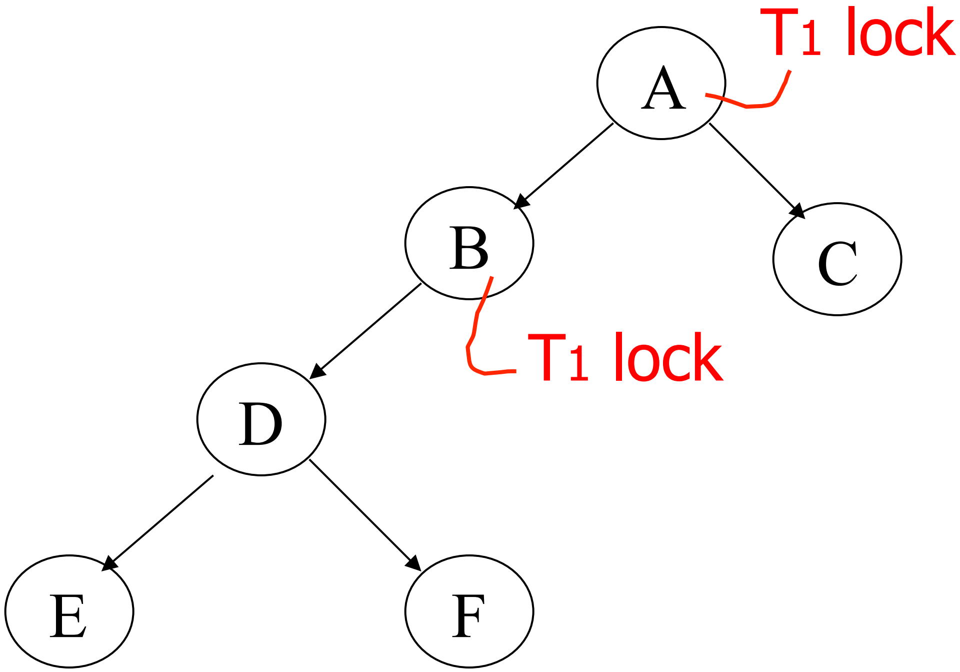

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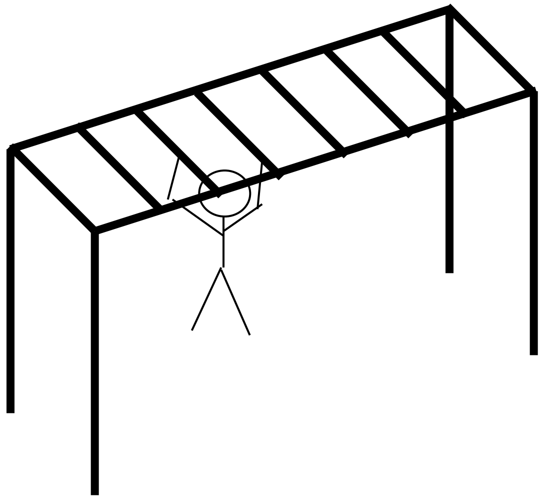
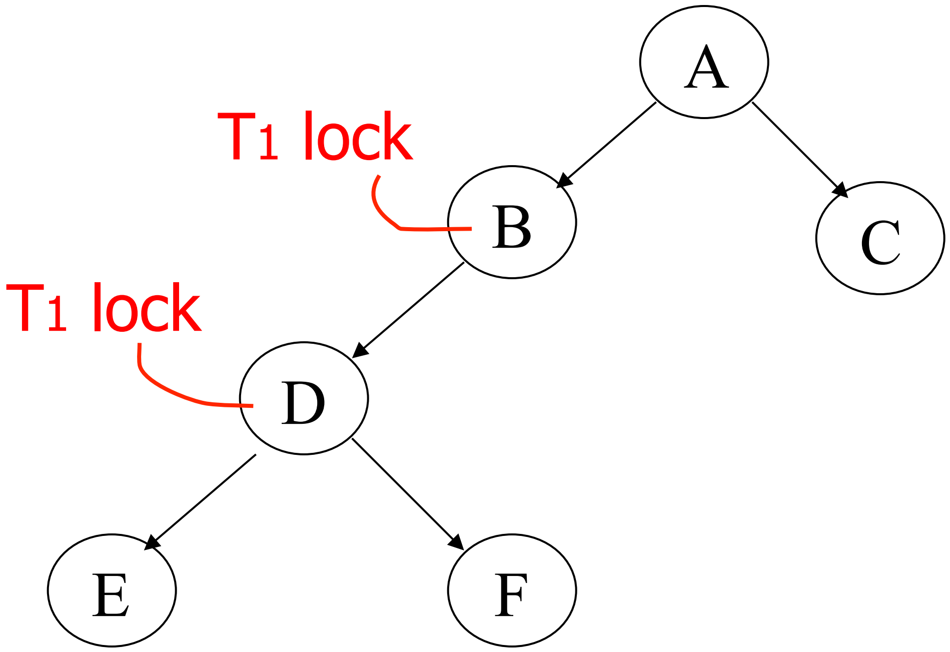


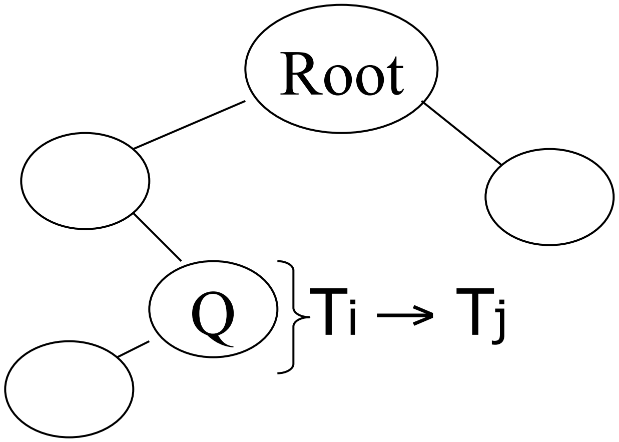


T$_1$ lock

T$_1$ lock

Why does this work?

- Assume all T_i start at root; exclusive lock
- $T_i \rightarrow T_j \Rightarrow T_i$ locks root before T_j



- Actually works if we don't always start at root

Rules: tree protocol (exclusive locks)

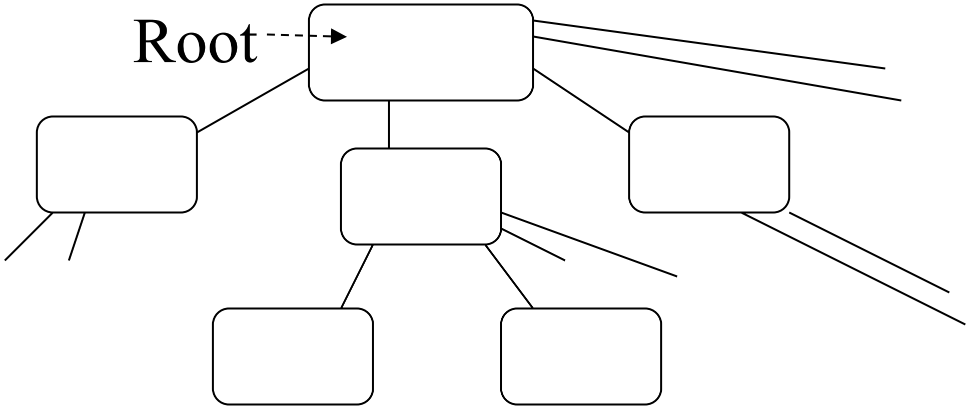
(1) First lock by T_i may be on any item

(2) After that, item Q can be locked by T_i only if parent(Q) locked by T_i

(3) Items may be unlocked at any time

(4) After T_i unlocks Q, it cannot relock Q

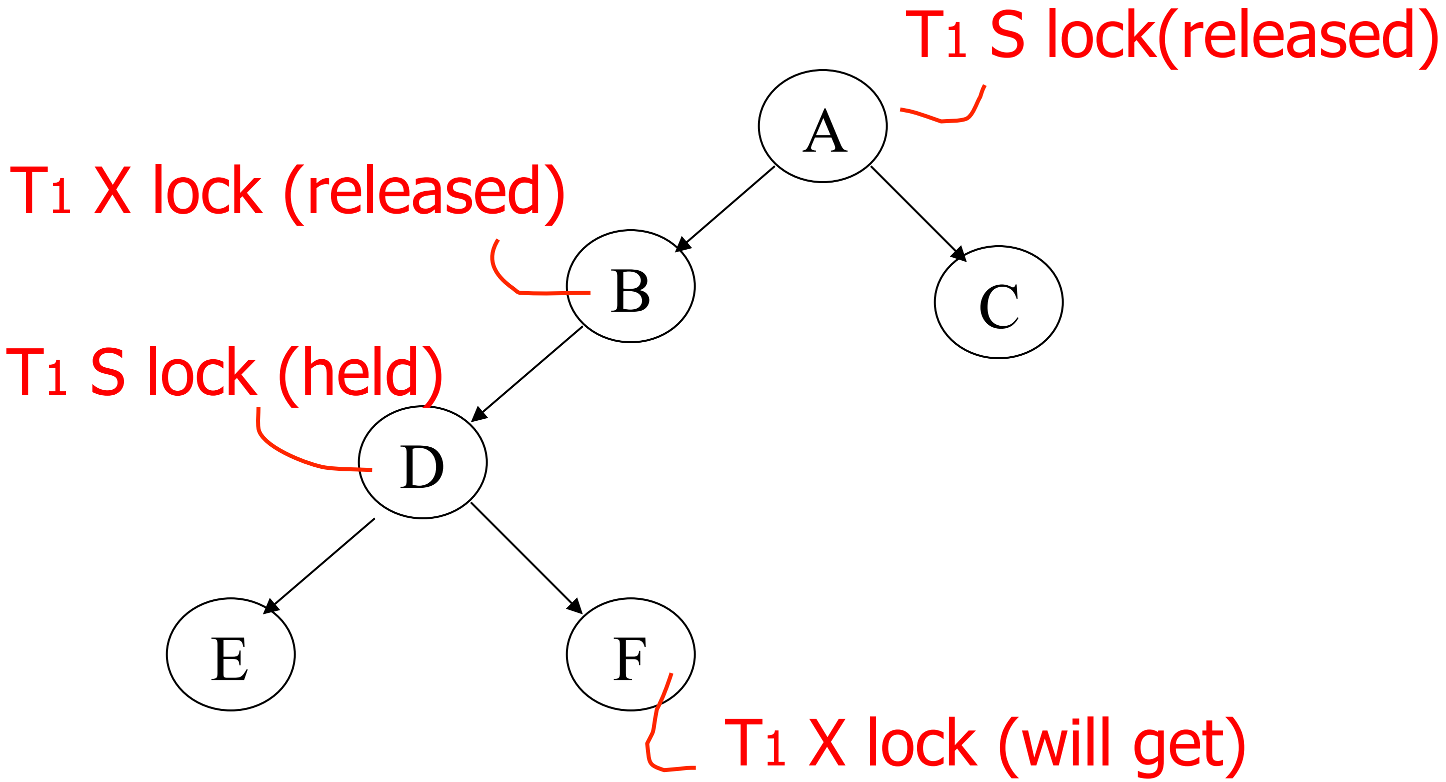
- Tree-like protocols are used typically for B-tree concurrency control



E.g., during insert, do not release parent lock, until you are certain child does not have to split

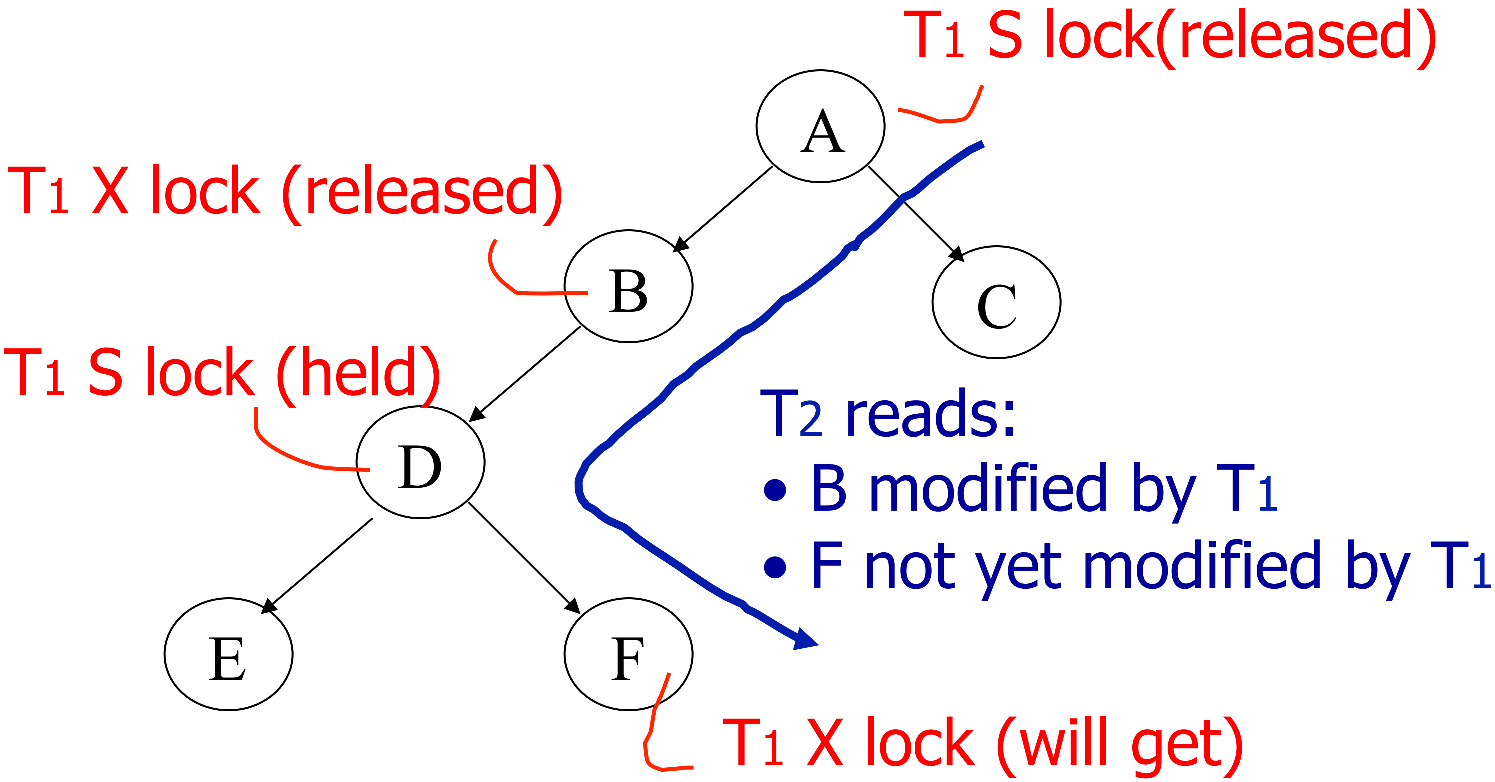
Tree Protocol with Shared Locks

- Rules for shared & exclusive locks?



Tree Protocol with Shared Locks

- Rules for shared & exclusive locks?



T_1 S lock(released)

T_1 X lock (released)

T_1 S lock (held)

T_1 X lock (will get)

T_2 reads:

- B modified by T_1
- F not yet modified by T_1

Tree Protocol with Shared Locks

- Need more restrictive protocol
- Will this work??
  - Once T_1 locks one object in X mode, all further locks down the tree must be in X mode

Validation

Transactions have 3 phases:

(1) Read

- all DB values read
- writes to temporary storage
- no locking

(2) Validate

- check if schedule so far is serializable

(3) Write

- if validate ok, write to DB

Key idea

- Make validation atomic
- If T_1, T_2, T_3, … is validation order, then resulting schedule will be conflict equivalent to $S_s = T_1\ T_2\ T_3$…

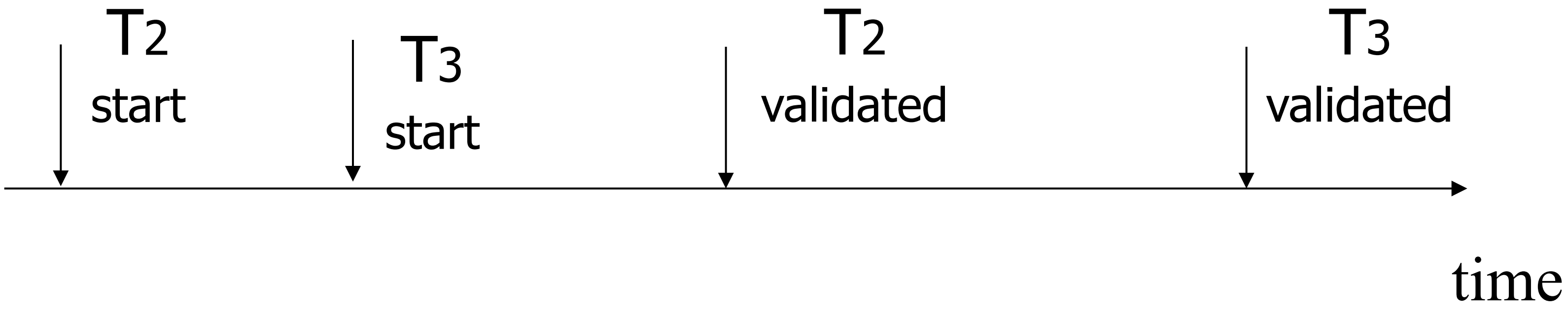
To implement validation, system keeps two sets:

- FIN = transactions that have finished phase 3 (and are all done)
- VAL = transactions that have successfully finished phase 2 (validation)

Example of what validation must prevent:

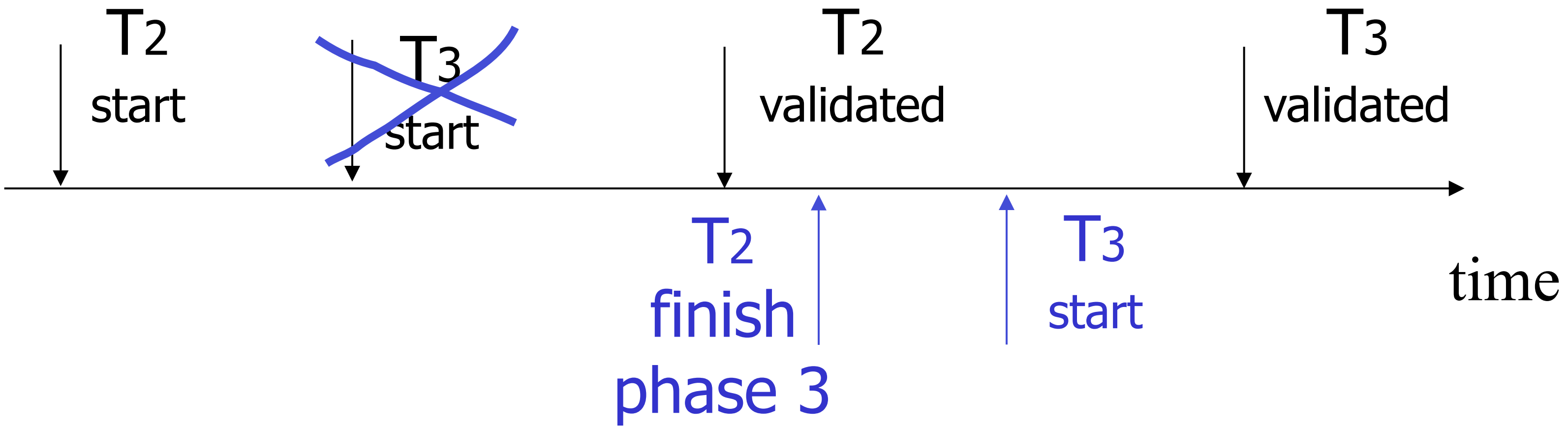
$RS(T_2)=\{B\} \qquad \cap \qquad RS(T_3)=\{A,B\} \neq \phi$

$WS(T_2)=\{B,D\} \qquad WS(T_3)=\{C\}$



Example of what validation must ~~prevent~~ allow:

$$\begin{array}{lcl} RS(T_2)=\{B\} & \cap & RS(T_3)=\{A,B\} \neq \phi \\ WS(T_2)=\{B,D\} & & WS(T_3)=\{C\} \end{array}$$



T2 start

~~T3 start~~

T2 validated

T3 validated

T2 finish phase 3

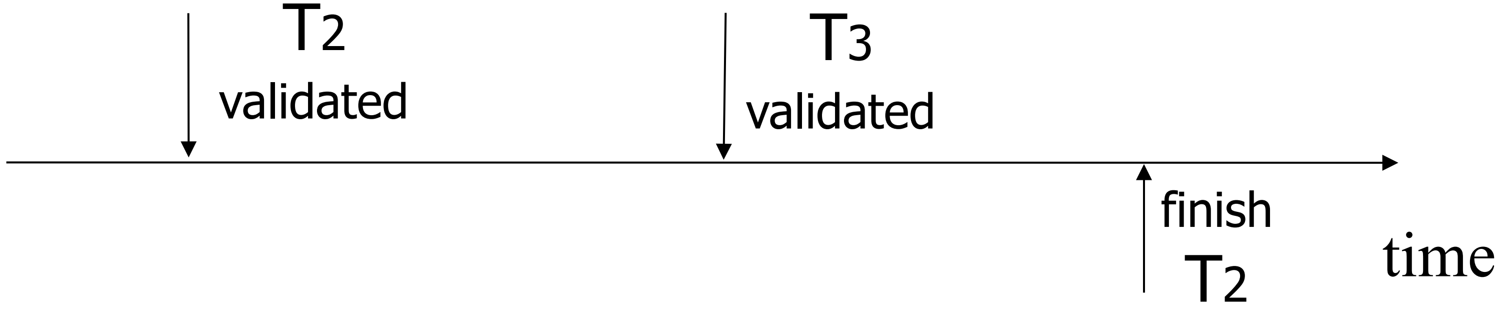
T3 start

time

Another thing validation must prevent:

$RS(T_2)=\{A\}$ $RS(T_3)=\{A,B\}$

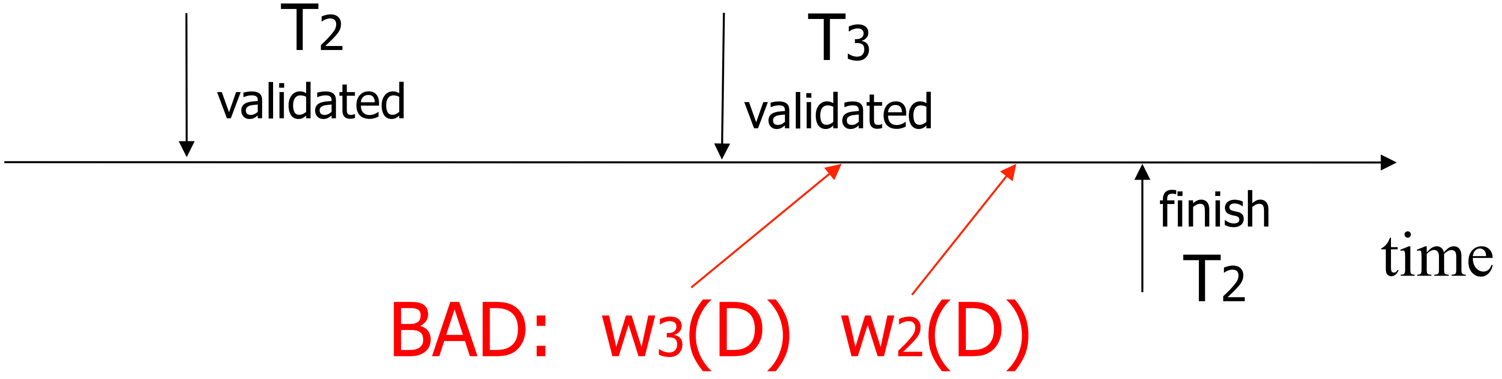
$WS(T_2)=\{D,E\}$ $WS(T_3)=\{C,D\}$



Another thing validation must prevent:

$RS(T_2)=\{A\}$ $RS(T_3)=\{A,B\}$

$WS(T_2)=\{D,E\}$ $WS(T_3)=\{C,D\}$



T_2 validated

T_3 validated

finish T_2

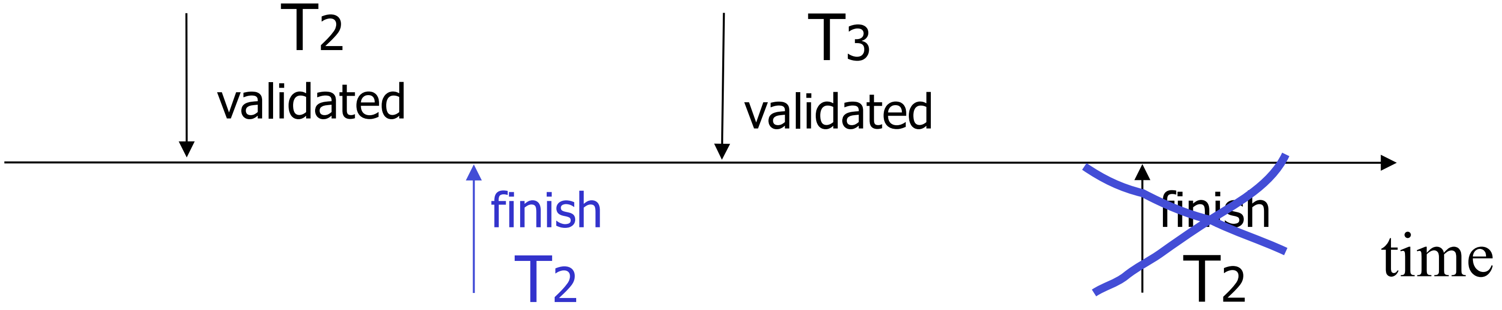
time

BAD: $w_3(D)$ $w_2(D)$

Another thing validation must ~~prevent~~ allow:

$RS(T_2)=\{A\}$ $\qquad$ $RS(T_3)=\{A,B\}$

$WS(T_2)=\{D,E\}$ $\qquad$ $WS(T_3)=\{C,D\}$



Validation rules for T_j:

```
(1) When $T_j$ starts phase 1:
        ignore($T_j$) $\leftarrow$ FIN
(2) at $T_j$ Validation:
                if check ($T_j$) then
                        [ VAL $\leftarrow$ VAL U {$T_j$};
                          do write phase;
                          FIN $\leftarrow$ FIN U {$T_j$} ]
```

```
Check ($T_j$):
    For $T_i \in$ VAL - IGNORE ($T_j$)  DO
        IF [ WS($T_i$) $\cap$ RS($T_j$) $\neq \emptyset$ OR
      $T_i \notin$ FIN ] THEN RETURN false;
    RETURN true;
```

```
Check ($T_j$):
    For $T_i \in$ VAL - IGNORE ($T_j$) DO
        IF [ $WS(T_i) \cap RS(T_j) \neq \emptyset$ OR
     $T_i \notin$ FIN ] THEN RETURN false;
    RETURN true;
```

Is this check too restrictive ?

Improving Check(T_j)

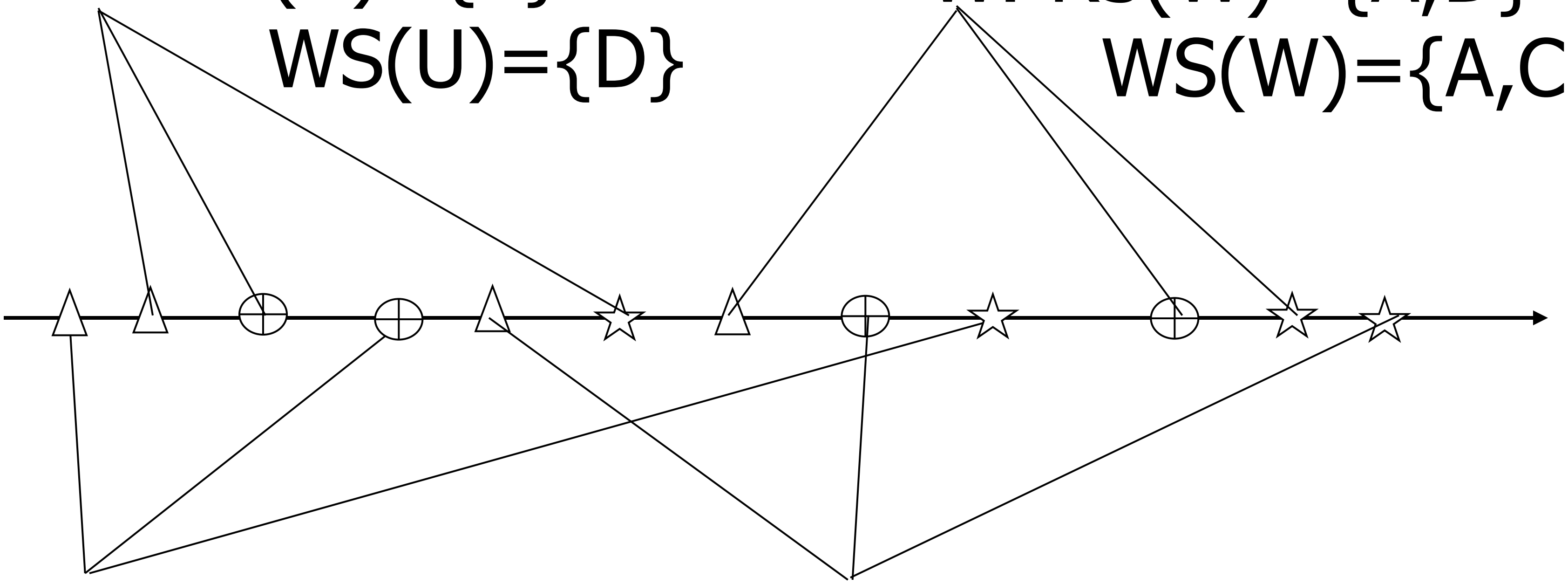
```
For $T_i \in$ VAL - IGNORE ($T_j$)  DO
  IF [ $WS(T_i) \cap RS(T_j) \neq \emptyset$ OR
      ($T_i \notin$ FIN  AND $WS(T_i) \cap WS(T_j) \neq \emptyset$)]
          THEN RETURN false;
RETURN true;
```

Exercise:

△ start
⊕ validate
☆ finish

U: RS(U)={B}
WS(U)={D}

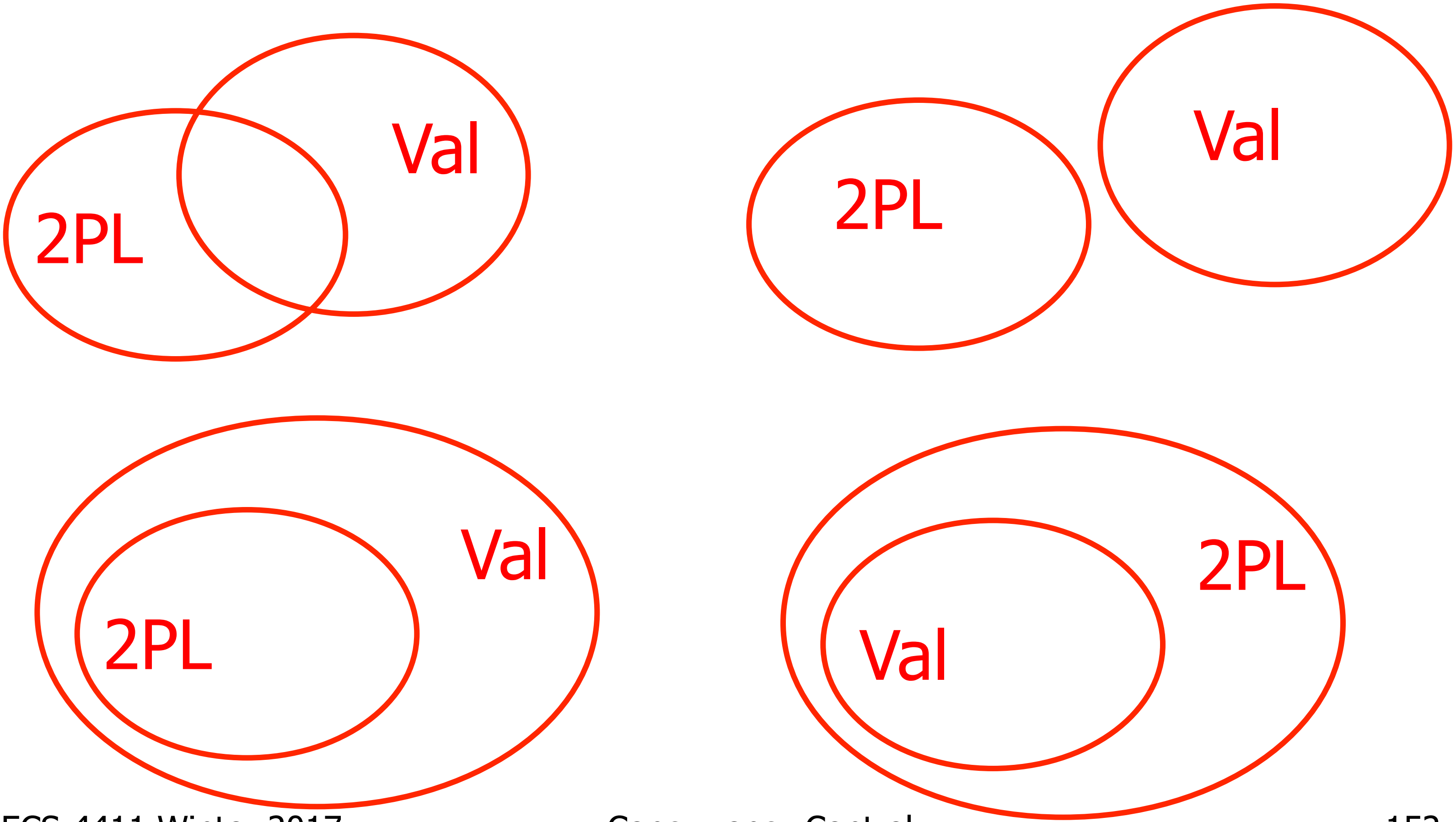
W: RS(W)={A,D}
WS(W)={A,C}



T: RS(T)={A,B}
WS(T)={A,C}

V: RS(V)={B}
WS(V)={D,E}

Is Validation = 2PL?



S2: w2(y) w1(x) w2(x)

- Achievable with 2PL?
- Achievable with validation?

S2: w2(y) w1(x) w2(x)

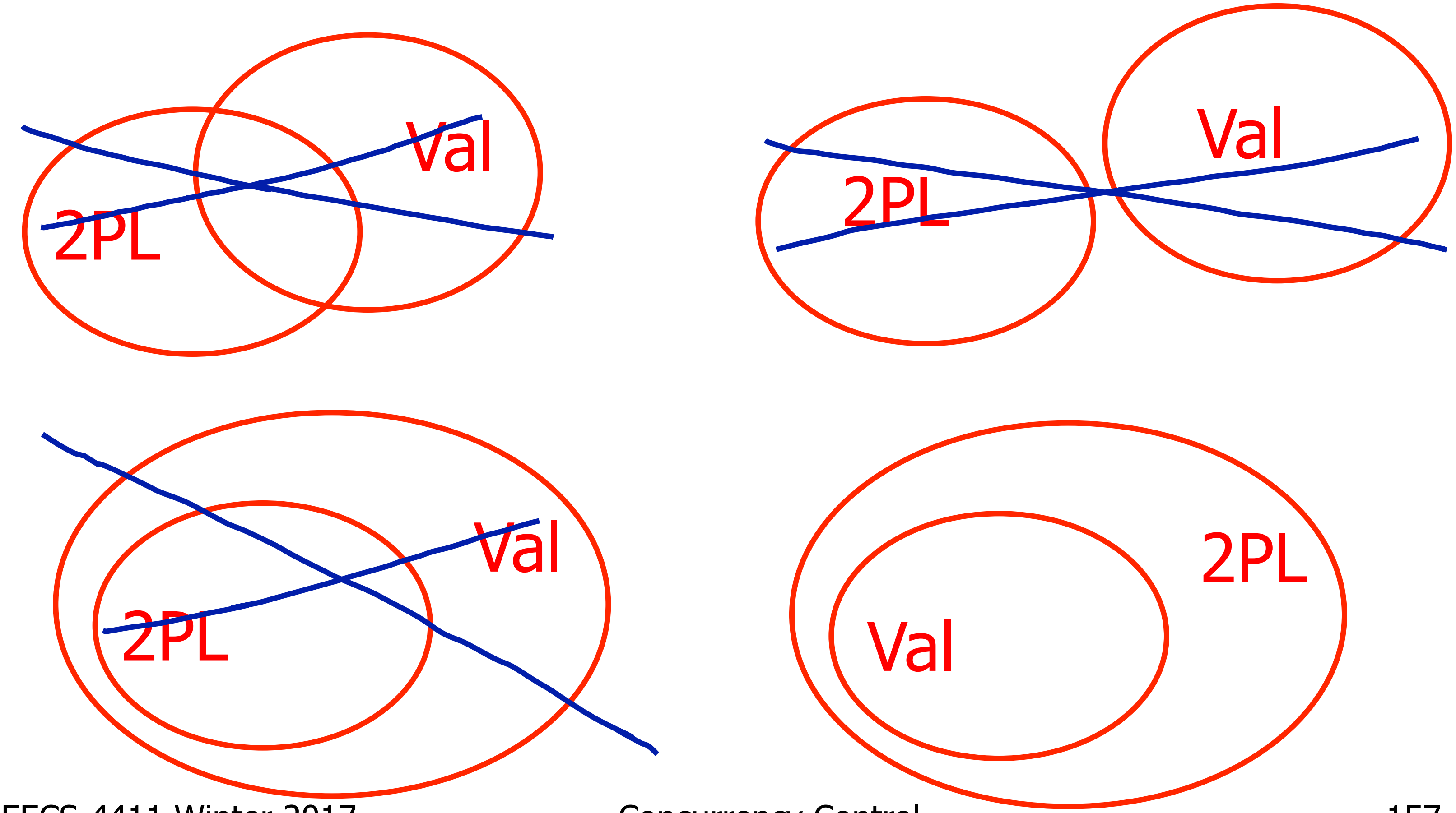
- S2 can be achieved with 2PL:
  l2(y) w2(y) l1(x) w1(x) u1(x) l2(x) w2(x) u2(y) u2(x)
- S2 cannot be achieved by validation:
  The validation point of T2, val2 must occur before w2(y) since transactions do not write to the database until after validation. Because of the conflict on x, val1 < val2, so we must have something like
  S2: val1 val2 w2(y) w1(x) w2(x)
  With the validation protocol, the writes of T2 should not start until T1 is all done with its writes, which is not the case.

Validation subset of 2PL?

- Possible proof (Check!):
  - Let S be validation schedule
  - For each T in S insert lock/unlocks, get S':
    - At T start: request read locks for all of RS(T)
    - At T validation: request write locks for WS(T); release read locks for read-only objects
    - At T end: release all write locks
  - Clearly transactions well-formed and 2PL
  - Must show S' is legal (next page)

- Say S′ not legal (due to w-r conflict):
  S′: ... l1(x)   w2(x) r1(x)   val1 u1(x) ...
  - At val1: T2 not in Ignore(T1); T2 in VAL
  - T1 does not validate: $WS(T2) \cap RS(T1) \neq \emptyset$
  - contradiction!
- Say S′ not legal (due to w-w conflict):
  S′: ... val1 l1(x)   w2(x) w1(x)   u1(x) ...
  - Say T2 validates first (proof similar if T1 validates first)
  - At val1: T2 not in Ignore(T1); T2 in VAL
  - T1 does not validate:
    $T2 \notin FIN$ AND $WS(T1) \cap WS(T2) \neq \emptyset$)
  - contradiction!

Conclusion:
Validation subset 2PL



Validation (also called optimistic concurrency control) is useful in some cases:

- Conflicts rare
- System resources plentiful
- Have real time constraints

Summary

Have studied C.C. mechanisms used in practice

- 2 PL

- Multiple granularity

- Tree (index) protocols

- Validation