



## 8. INTRACTABILITY I

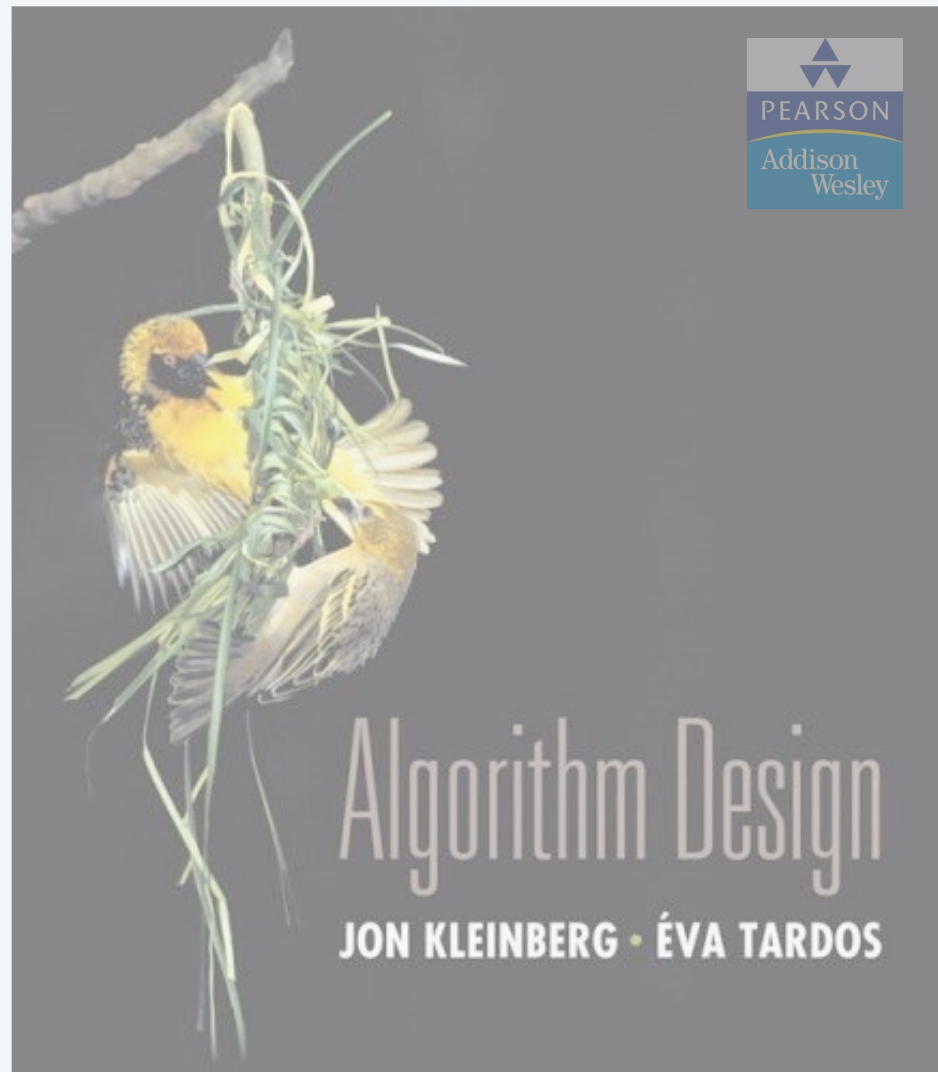
---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ *numerical problems*

Lecture slides by Kevin Wayne

Copyright © 2005 Pearson–Addison Wesley

<http://www.cs.princeton.edu/~wayne/kleinberg-tardos>



## 8. INTRACTABILITY I

---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ *numerical problems*

# Algorithm design patterns and antipatterns

---

## Algorithm design patterns.

- Greedy.
- Divide and conquer.
- Dynamic programming.
- Duality.
- **Reductions.**
- Local search.
- Randomization.

## Algorithm design antipatterns.

- **NP-completeness.**  $O(n^k)$  algorithm unlikely.
- PSPACE-completeness.  $O(n^k)$  certification algorithm unlikely.
- Undecidability. No algorithm possible.

# Classify problems according to computational requirements

---

Q. Which problems will we be able to solve in practice?

A working definition. Those with polynomial-time algorithms.



von Neumann  
(1953)



Nash  
(1955)



Gödel  
(1956)



Cobham  
(1964)



Edmonds  
(1965)



Rabin  
(1966)

Theory. Definition is broad and robust.

Practice. Poly-time algorithms scale to huge problems.

constants tend to be small, e.g.,  $3n^2$



# Classify problems according to computational requirements

---

Q. Which problems will we be able to solve in practice?

A working definition. Those with polynomial-time algorithms.

yes	probably no
shortest path	longest path
min cut	max cut
2-satisfiability	3-satisfiability
planar 4-colorability	planar 3-colorability
bipartite vertex cover	vertex cover
matching	3d-matching
primality testing	factoring
linear programming	integer linear programming

# Classify problems

---

**Desiderata.** Classify problems according to those that can be solved in polynomial time and those that cannot.

**Provably requires exponential time.**

- Given a constant-size program, does it halt in at most  $k$  steps?
- Given a board position in an  $n$ -by- $n$  generalization of checkers, can black guarantee a win?

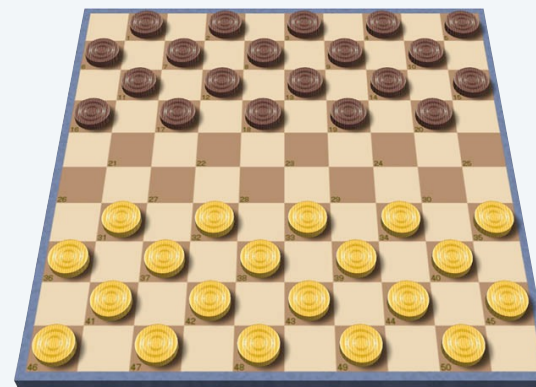
input size =  $c + \lg k$



using forced capture rule



*Alan designed the perfect computer*



**Frustrating news.** Huge number of fundamental problems have defied classification for decades.

# Polynomial-time reductions

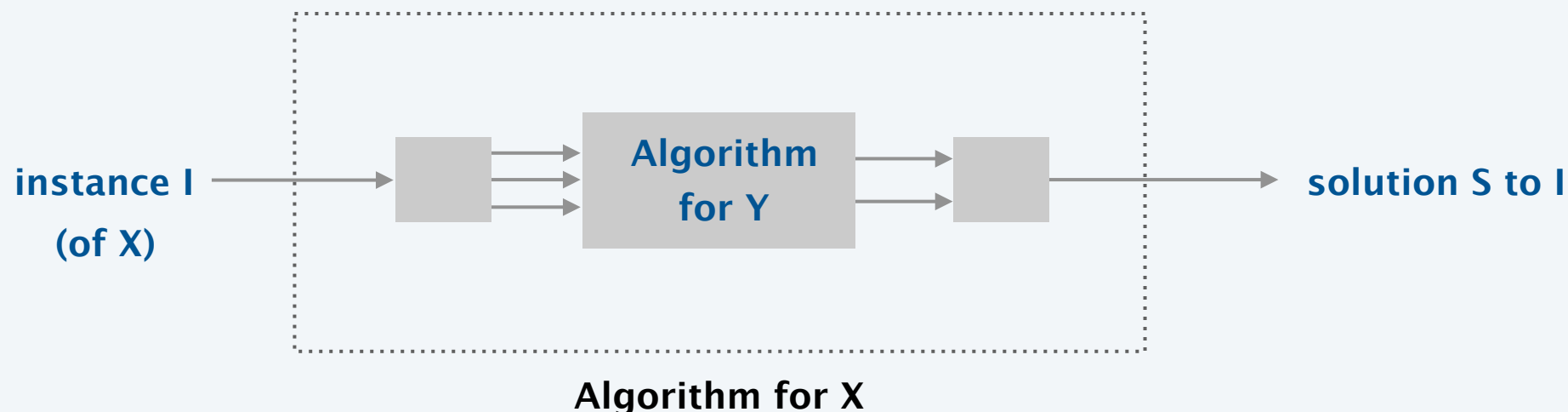
---

**Desiderata'.** Suppose we could solve problem  $Y$  in polynomial-time. What else could we solve in polynomial time?

**Reduction.** Problem  $X$  **polynomial-time (Cook) reduces to** problem  $Y$  if arbitrary instances of problem  $X$  can be solved using:

- Polynomial number of standard computational steps, plus
- Polynomial number of calls to oracle that solves problem  $Y$ .

computational model supplemented by special piece of hardware that solves instances of  $Y$  in a single step



# Polynomial-time reductions

---

**Desiderata'.** Suppose we could solve problem  $Y$  in polynomial-time. What else could we solve in polynomial time?

**Reduction.** Problem  $X$  **polynomial-time (Cook) reduces to** problem  $Y$  if arbitrary instances of problem  $X$  can be solved using:

- Polynomial number of standard computational steps, plus
- Polynomial number of calls to oracle that solves problem  $Y$ .

**Notation.**  $X \leq_p Y$ .

**Note.** We pay for time to write down instances sent to oracle  $\Rightarrow$  instances of  $Y$  must be of polynomial size.

**Caveat.** Don't mistake  $X \leq_p Y$  with  $Y \leq_p X$ .



# Polynomial-time reductions

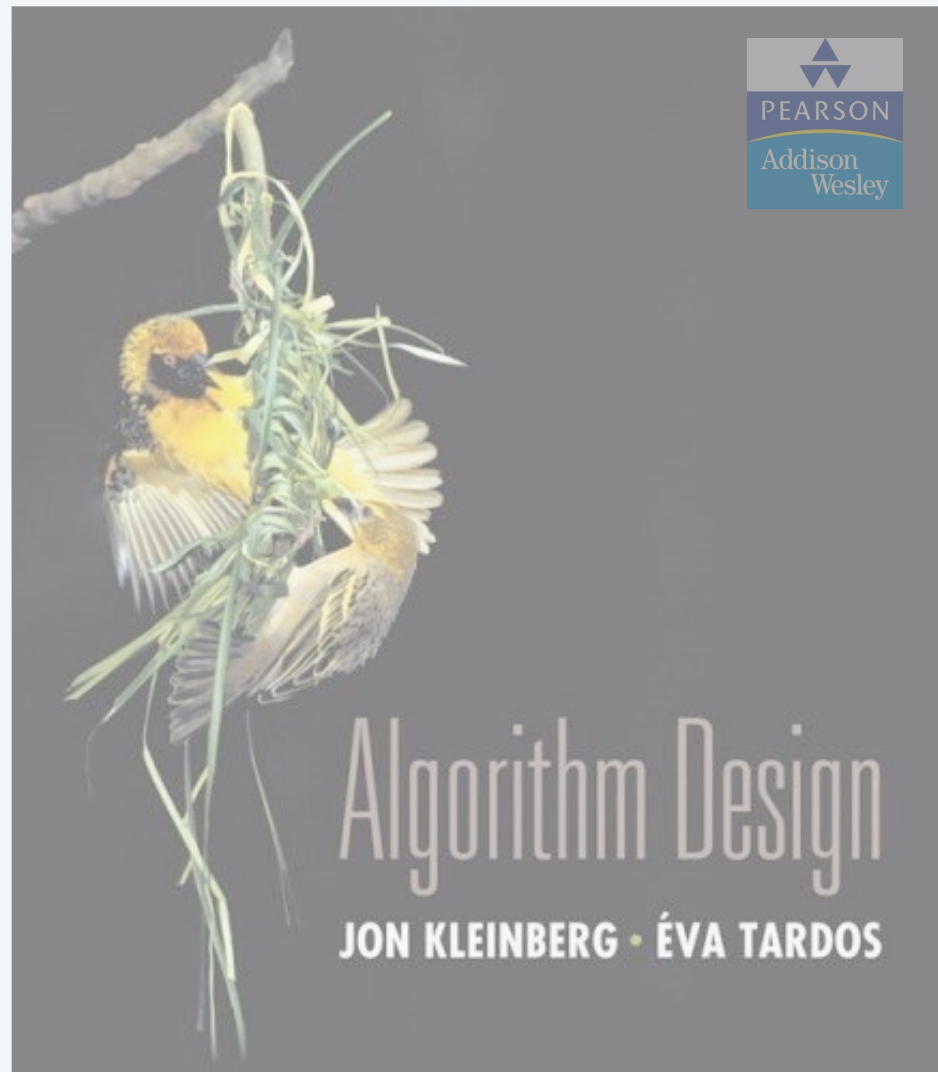
---

**Design algorithms.** If  $X \leq_p Y$  and  $Y$  can be solved in polynomial time, then  $X$  can be solved in polynomial time.

**Establish intractability.** If  $X \leq_p Y$  and  $X$  cannot be solved in polynomial time, then  $Y$  cannot be solved in polynomial time.

**Establish equivalence.** If both  $X \leq_p Y$  and  $Y \leq_p X$ , we use notation  $X \equiv_p Y$ . In this case,  $X$  can be solved in polynomial time iff  $Y$  can be.

**Bottom line.** Reductions classify problems according to **relative** difficulty.



## 8. INTRACTABILITY I

---

- ▶ *poly-time reductions*
- ▶ ***packing and covering problems***
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ *numerical problems*

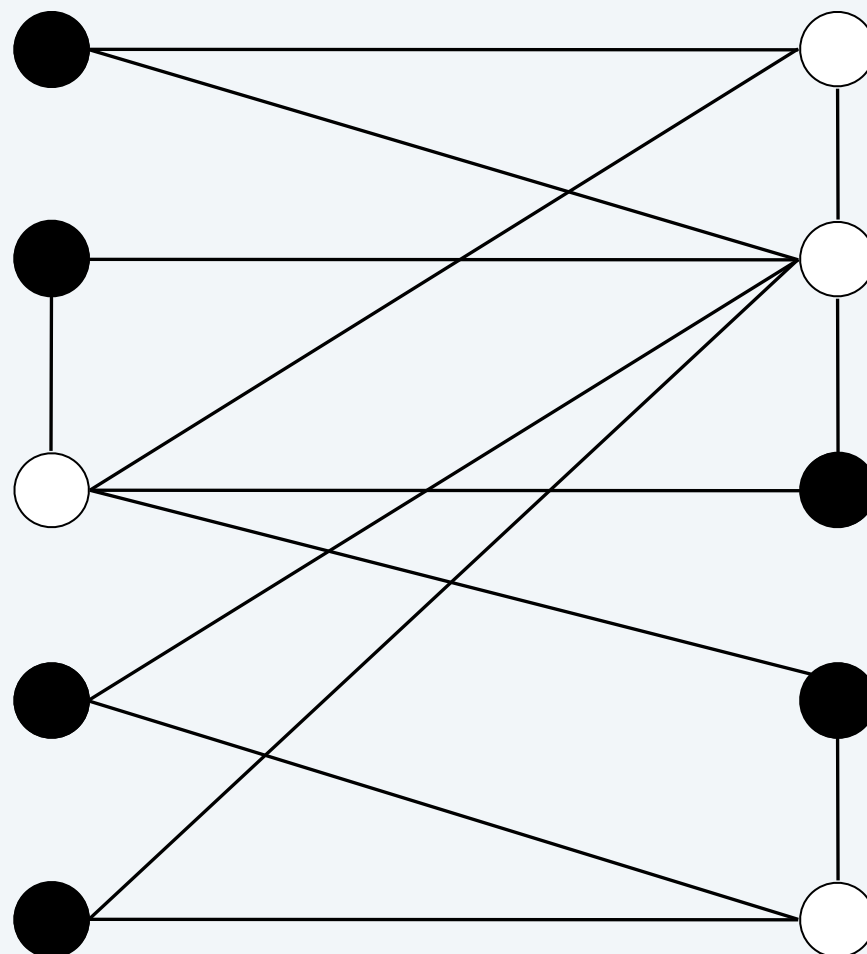
# Independent set

---

**INDEPENDENT-SET.** Given a graph  $G = (V, E)$  and an integer  $k$ , is there a subset of vertices  $S \subseteq V$  such that  $|S| \geq k$ , and for each edge at most one of its endpoints is in  $S$ ?

**Ex.** Is there an independent set of size  $\geq 6$ ?

**Ex.** Is there an independent set of size  $\geq 7$ ?



● independent set of size 6

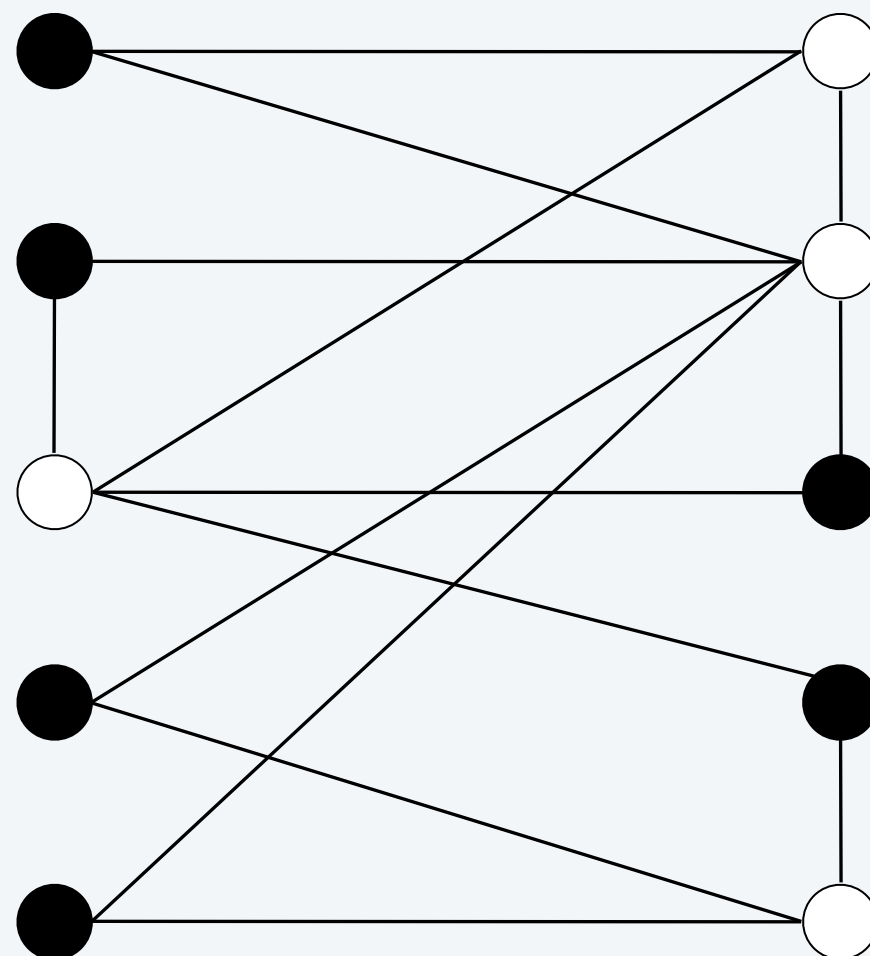
# Vertex cover

---

**VERTEX-COVER.** Given a graph  $G = (V, E)$  and an integer  $k$ , is there a subset of vertices  $S \subseteq V$  such that  $|S| \leq k$ , and for each edge, at least one of its endpoints is in  $S$ ?

**Ex.** Is there a vertex cover of size  $\leq 4$ ?

**Ex.** Is there a vertex cover of size  $\leq 3$ ?



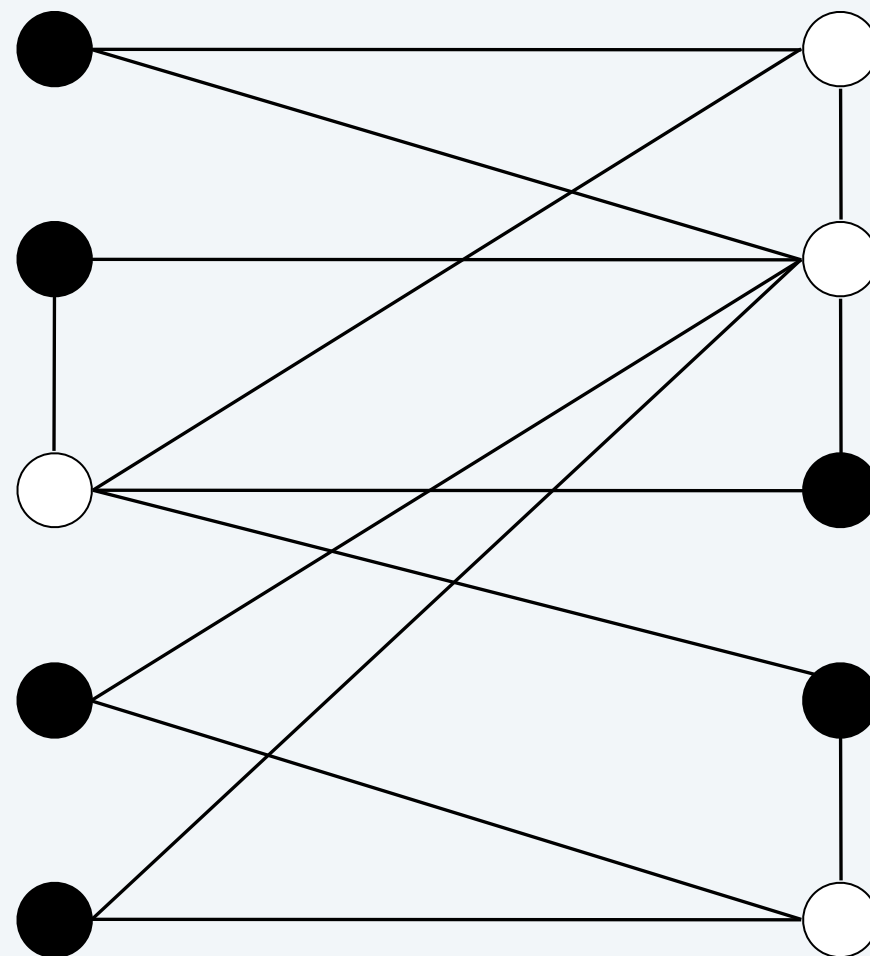
● independent set of size 6  
○ vertex cover of size 4

# Vertex cover and independent set reduce to one another

---

**Theorem.** VERTEX-COVER  $\equiv_P$  INDEPENDENT-SET.

**Pf.** We show  $S$  is an independent set of size  $k$  iff  $V - S$  is a vertex cover of size  $n - k$ .



● independent set of size 6  
○ vertex cover of size 4

# Vertex cover and independent set reduce to one another

---

**Theorem.** VERTEX-COVER  $\equiv_P$  INDEPENDENT-SET.

**Pf.** We show  $S$  is an independent set of size  $k$  iff  $V - S$  is a vertex cover of size  $n - k$ .

$\Rightarrow$

- Let  $S$  be any independent set of size  $k$ .
- $V - S$  is of size  $n - k$ .
- Consider an arbitrary edge  $(u, v)$ .
- $S$  independent  $\Rightarrow$  either  $u \notin S$  or  $v \notin S$  (or both)  
 $\Rightarrow$  either  $u \in V - S$  or  $v \in V - S$  (or both).
- Thus,  $V - S$  covers  $(u, v)$ .

# Vertex cover and independent set reduce to one another

---

**Theorem.** VERTEX-COVER  $\equiv_P$  INDEPENDENT-SET.

**Pf.** We show  $S$  is an independent set of size  $k$  iff  $V - S$  is a vertex cover of size  $n - k$ .

$\Leftarrow$

- Let  $V - S$  be any vertex cover of size  $n - k$ .
- $S$  is of size  $k$ .
- Consider two nodes  $u \in S$  and  $v \in S$ .
- Observe that  $(u, v) \notin E$  since  $V - S$  is a vertex cover.
- Thus, no two nodes in  $S$  are joined by an edge  $\Rightarrow S$  independent set. ■

# Set cover

---

**SET-COVER.** Given a set  $U$  of elements, a collection  $S$  of subsets of  $U$ , and an integer  $k$ , are there  $\leq k$  of these subsets whose union is equal to  $U$ ?

## Sample application.

- $m$  available pieces of software.
- Set  $U$  of  $n$  capabilities that we would like our system to have.
- The  $i^{th}$  piece of software provides the set  $S_i \subseteq U$  of capabilities.
- Goal: achieve all  $n$  capabilities using fewest pieces of software.

$$U = \{ 1, 2, 3, 4, 5, 6, 7 \}$$

$$S_a = \{ 3, 7 \}$$

$$S_b = \{ 2, 4 \}$$

$$S_c = \{ 3, 4, 5, 6 \}$$

$$S_d = \{ 5 \}$$

$$S_e = \{ 1 \}$$

$$S_f = \{ 1, 2, 6, 7 \}$$

$$k = 2$$

a set cover instance



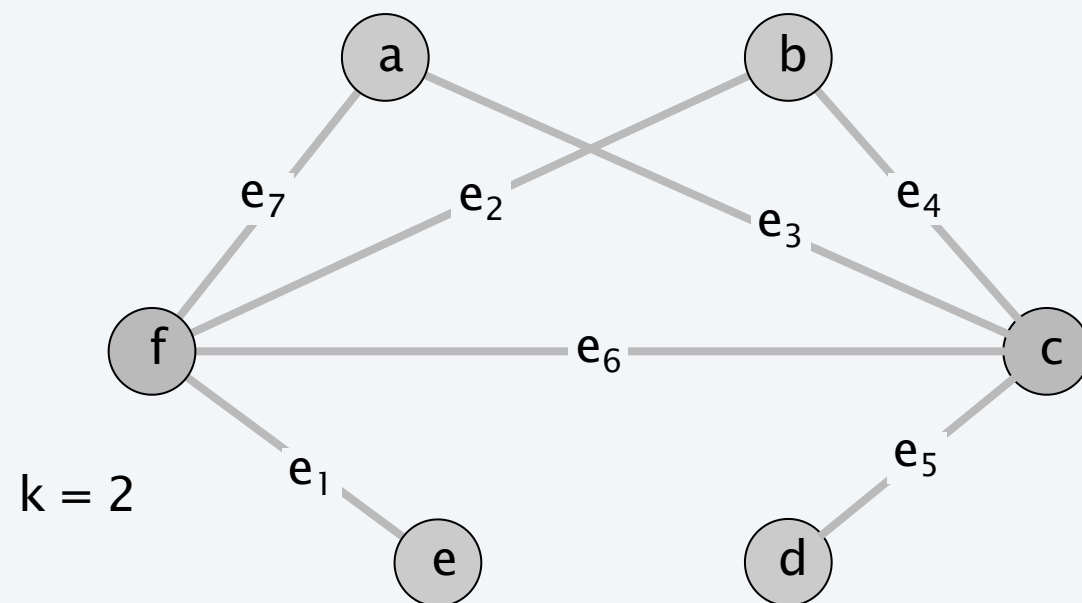
# Vertex cover reduces to set cover

**Theorem.** VERTEX-COVER  $\leq_P$  SET-COVER.

**Pf.** Given a VERTEX-COVER instance  $G = (V, E)$  and  $k$ , we construct a SET-COVER instance  $(U, S, k)$  that has a set cover of size  $k$  iff  $G$  has a vertex cover of size  $k$ .

**Construction.**

- Universe  $U = E$ .
- Include one subset for each node  $v \in V$ :  $S_v = \{e \in E : e \text{ incident to } v\}$ .



vertex cover instance  
( $k = 2$ )

$U = \{ 1, 2, 3, 4, 5, 6, 7 \}$	
$S_a = \{ 3, 7 \}$	$S_b = \{ 2, 4 \}$
$S_c = \{ 3, 4, 5, 6 \}$	$S_d = \{ 5 \}$
$S_e = \{ 1 \}$	$S_f = \{ 1, 2, 6, 7 \}$

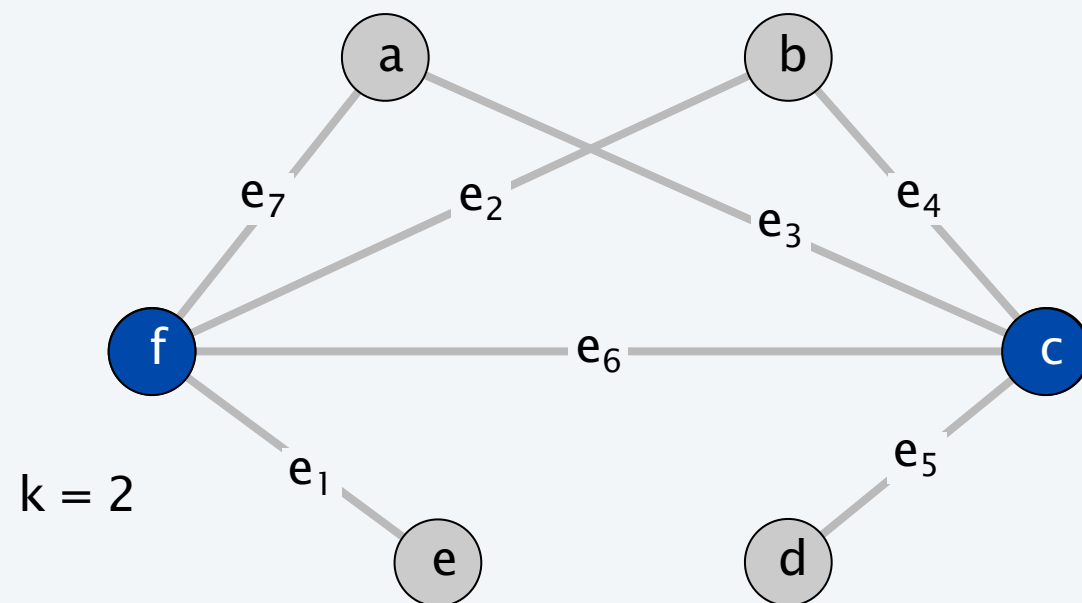
set cover instance  
( $k = 2$ )

# Vertex cover reduces to set cover

**Lemma.**  $G = (V, E)$  contains a vertex cover of size  $k$  iff  $(U, S, k)$  contains a set cover of size  $k$ .

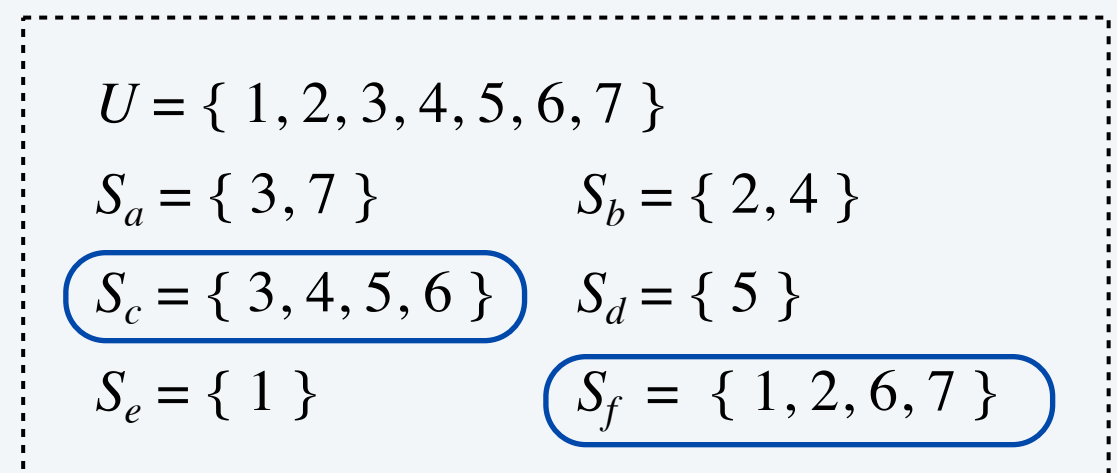
**Pf.**  $\Rightarrow$  Let  $X \subseteq V$  be a vertex cover of size  $k$  in  $G$ .

- Then  $Y = \{ S_v : v \in X \}$  is a set cover of size  $k$ . ■



$k = 2$

vertex cover instance  
( $k = 2$ )



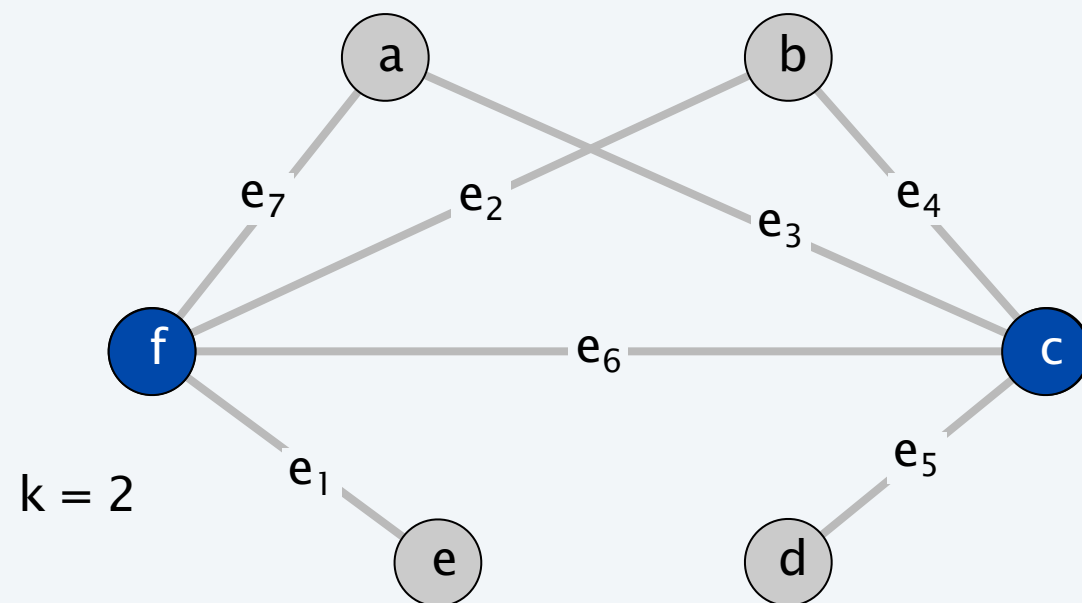
set cover instance  
( $k = 2$ )

# Vertex cover reduces to set cover

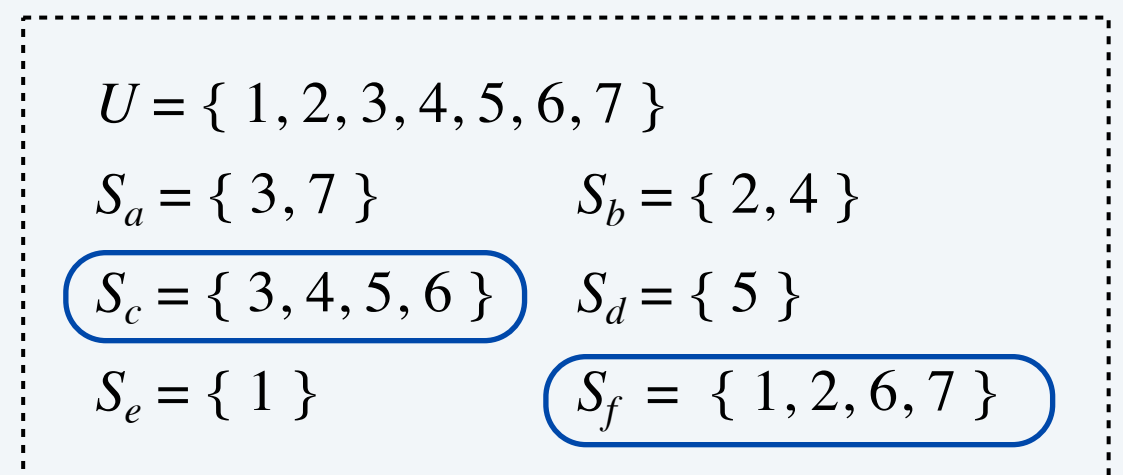
**Lemma.**  $G = (V, E)$  contains a vertex cover of size  $k$  iff  $(U, S, k)$  contains a set cover of size  $k$ .

**Pf.**  $\Leftarrow$  Let  $Y \subseteq S$  be a set cover of size  $k$  in  $(U, S, k)$ .

- Then  $X = \{ v : S_v \in Y \}$  is a vertex cover of size  $k$  in  $G$ . ■



vertex cover instance  
( $k = 2$ )



set cover instance  
( $k = 2$ )



## 8. INTRACTABILITY I

---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ ***constraint satisfaction problems***
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ *numerical problems*

# Satisfiability

---

**Literal.** A Boolean variable or its negation.

$$x_i \text{ or } \overline{x_i}$$

**Clause.** A disjunction of literals.

$$C_j = x_1 \vee \overline{x_2} \vee x_3$$

**Conjunctive normal form (CNF).** A propositional formula  $\Phi$  that is a conjunction of clauses.

$$\Phi = C_1 \wedge C_2 \wedge C_3 \wedge C_4$$

**SAT.** Given a CNF formula  $\Phi$ , does it have a satisfying truth assignment?

**3-SAT.** SAT where each clause contains exactly 3 literals (and each literal corresponds to a different variable).

$$\Phi = (\overline{x_1} \vee x_2 \vee x_3) \wedge (x_1 \vee \overline{x_2} \vee x_3) \wedge (\overline{x_1} \vee x_2 \vee x_4)$$

**yes instance:**  $x_1 = \text{true}$ ,  $x_2 = \text{true}$ ,  $x_3 = \text{false}$ ,  $x_4 = \text{false}$

**Key application.** Electronic design automation (EDA).

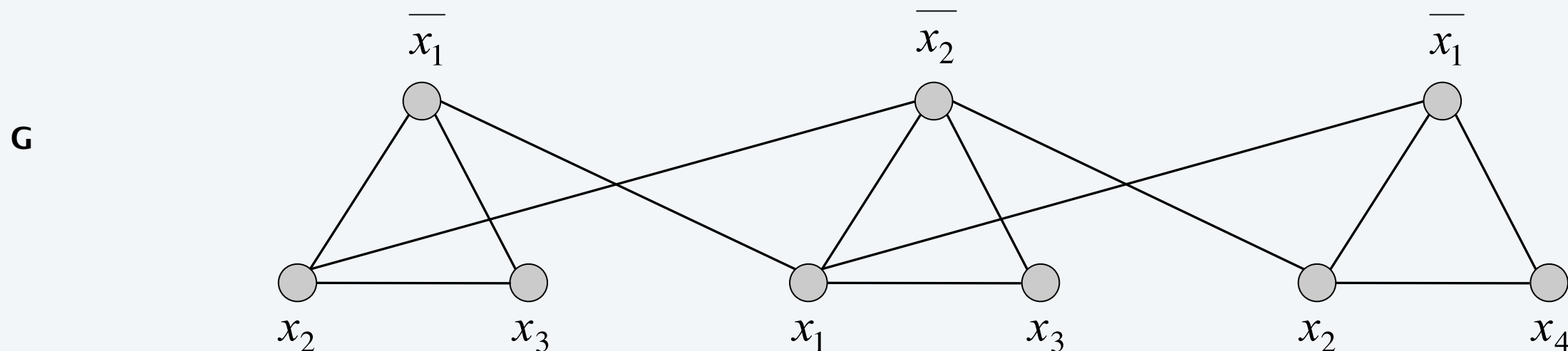
# 3-satisfiability reduces to independent set

**Theorem.**  $3\text{-SAT} \leq_P \text{INDEPENDENT-SET}$ .

**Pf.** Given an instance  $\Phi$  of 3-SAT, we construct an instance  $(G, k)$  of INDEPENDENT-SET that has an independent set of size  $k = |\Phi|$  iff  $\Phi$  is satisfiable.

**Construction.**

- $G$  contains 3 nodes for each clause, one for each literal.
- Connect 3 literals in a clause in a triangle.
- Connect literal to each of its negations.



**k = 3**

$$\Phi = (\bar{x}_1 \vee x_2 \vee x_3) \wedge (x_1 \vee \bar{x}_2 \vee x_3) \wedge (\bar{x}_1 \vee x_2 \vee x_4)$$

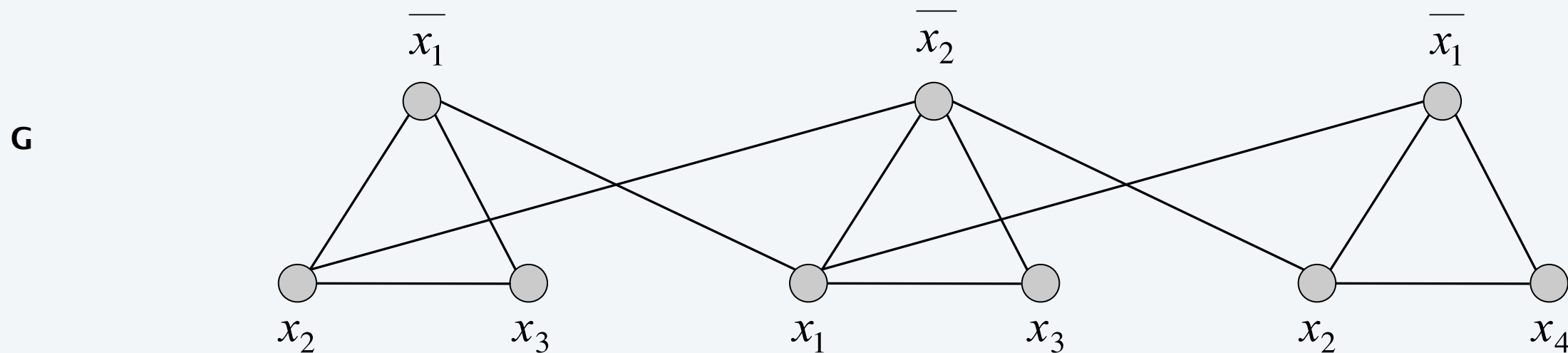
# 3-satisfiability reduces to independent set

**Lemma.**  $G$  contains independent set of size  $k = |\Phi|$  iff  $\Phi$  is satisfiable.

**Pf.**  $\Rightarrow$  Let  $S$  be independent set of size  $k$ .

- $S$  must contain exactly one node in each triangle.
- Set these literals to *true* (and remaining variables consistently).
- Truth assignment is consistent and all clauses are satisfied.

**Pf**  $\Leftarrow$  Given satisfying assignment, select one true literal from each triangle. This is an independent set of size  $k$ . ■



**k = 3**

$$\Phi = \left( \overline{x_1} \vee x_2 \vee x_3 \right) \wedge \left( x_1 \vee \overline{x_2} \vee x_3 \right) \wedge \left( \overline{x_1} \vee x_2 \vee x_4 \right)$$

# Review

---

## Basic reduction strategies.

- Simple equivalence:  $\text{INDEPENDENT-SET} \equiv_p \text{VERTEX-COVER}$ .
- Special case to general case:  $\text{VERTEX-COVER} \leq_p \text{SET-COVER}$ .
- Encoding with gadgets:  $3\text{-SAT} \leq_p \text{INDEPENDENT-SET}$ .

**Transitivity.** If  $X \leq_p Y$  and  $Y \leq_p Z$ , then  $X \leq_p Z$ .

**Pf idea.** Compose the two algorithms.

**Ex.**  $3\text{-SAT} \leq_p \text{INDEPENDENT-SET} \leq_p \text{VERTEX-COVER} \leq_p \text{SET-COVER}$ .



# Search problems


---

**Decision problem.** Does there **exist** a vertex cover of size  $\leq k$ ?

**Search problem.** **Find** a vertex cover of size  $\leq k$ .

**Ex.** To find a vertex cover of size  $\leq k$  :

- Determine if there exists a vertex cover of size  $\leq k$ .
- Find a vertex  $v$  such that  $G - \{v\}$  has a vertex cover of size  $\leq k - 1$ .  
(any vertex in any vertex cover of size  $\leq k$  will have this property)
- Include  $v$  in the vertex cover.
- Recursively find a vertex cover of size  $\leq k - 1$  in  $G - \{v\}$ .

 delete  $v$  and all incident edges

**Bottom line.** VERTEX-COVER  $\equiv_P$  FIND-VERTEX-COVER.

# Optimization problems

---

**Decision problem.** Does there **exist** a vertex cover of size  $\leq k$ ?

**Search problem.** **Find** a vertex cover of size  $\leq k$ .

**Optimization problem.** **Find** a vertex cover of **minimum** size.

**Ex.** To find vertex cover of minimum size:

- (Binary) search for size  $k^*$  of min vertex cover.
- Solve corresponding search problem.

**Bottom line.** VERTEX-COVER  $\equiv_P$  FIND-VERTEX-COVER  $\equiv_P$  OPTIMAL-VERTEX-COVER.



## 8. INTRACTABILITY I

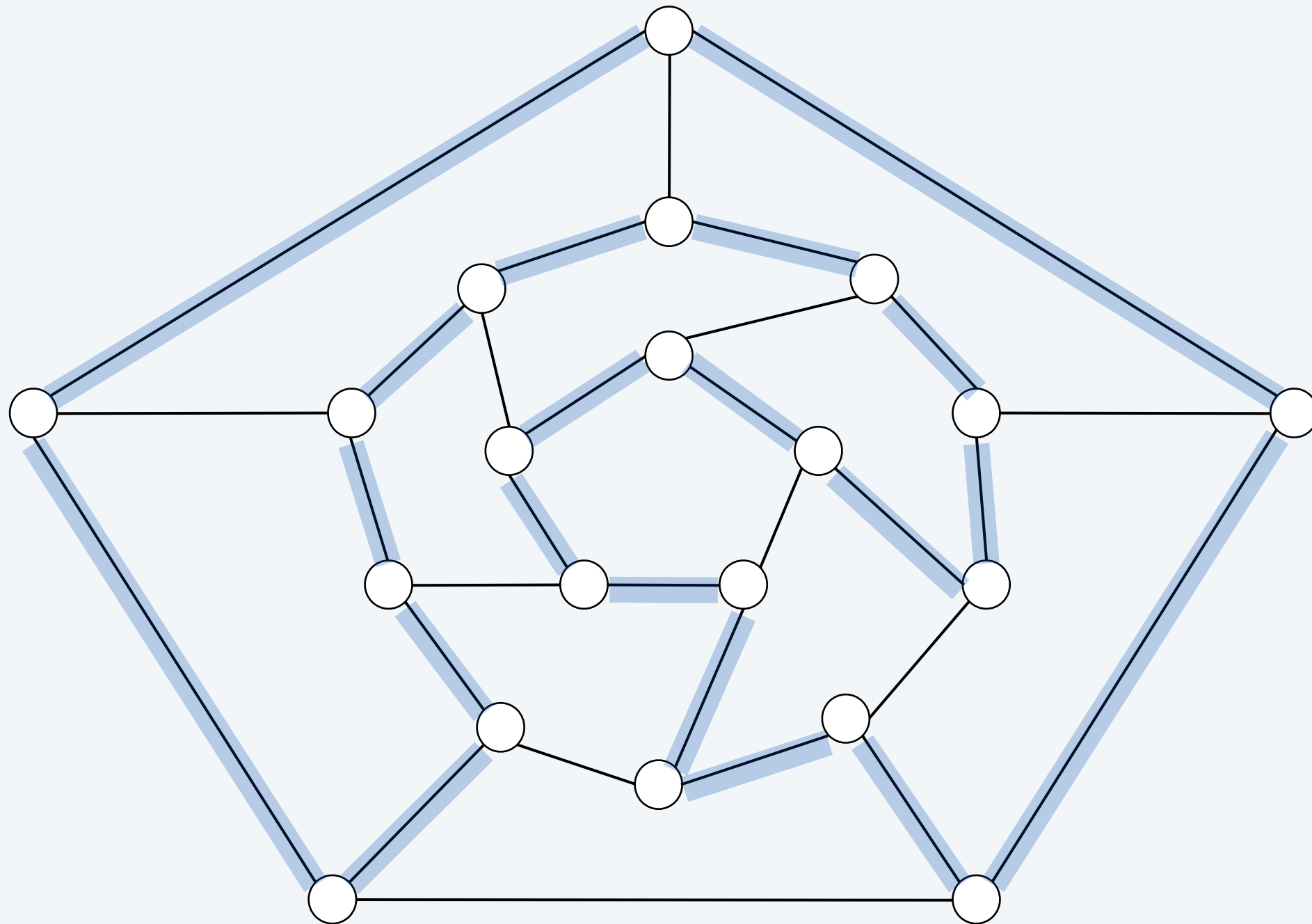
---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ ***sequencing problems***
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ *numerical problems*

# Hamilton cycle

---

**HAM-CYCLE.** Given an undirected graph  $G = (V, E)$ , does there exist a simple cycle  $\Gamma$  that contains every node in  $V$ ?

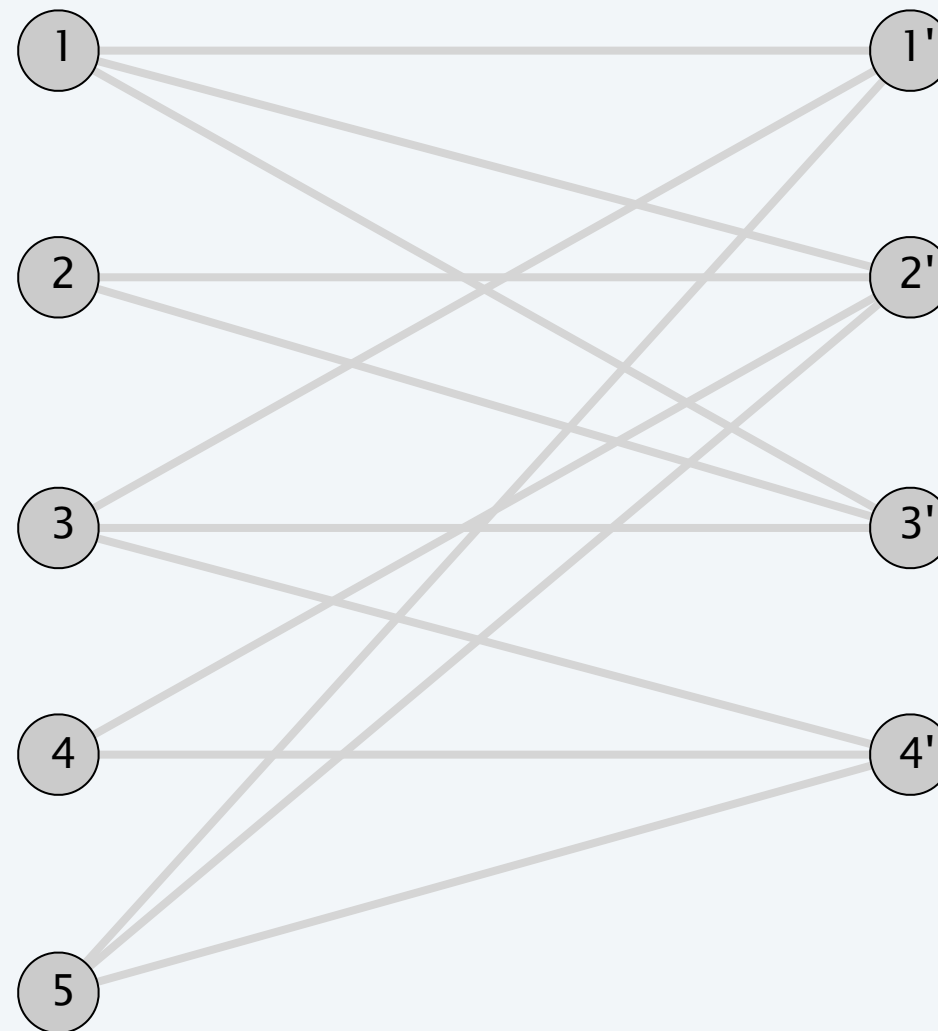


yes

# Hamilton cycle

---

**HAM-CYCLE.** Given an undirected graph  $G = (V, E)$ , does there exist a simple cycle  $\Gamma$  that contains every node in  $V$ ?



**no**

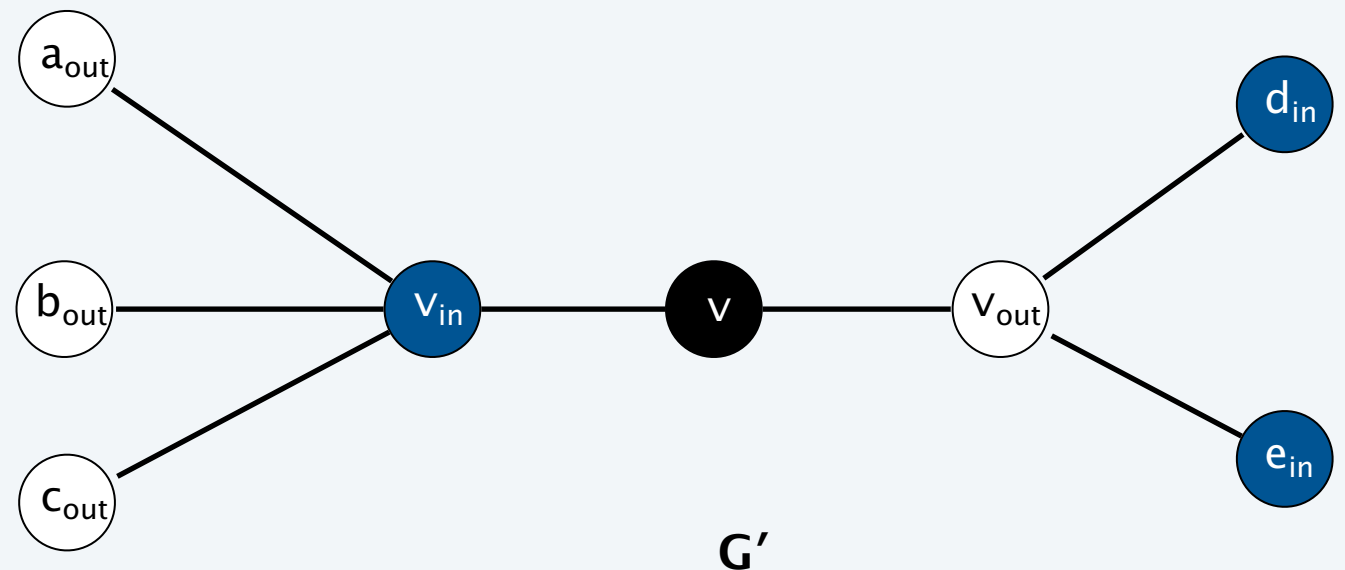
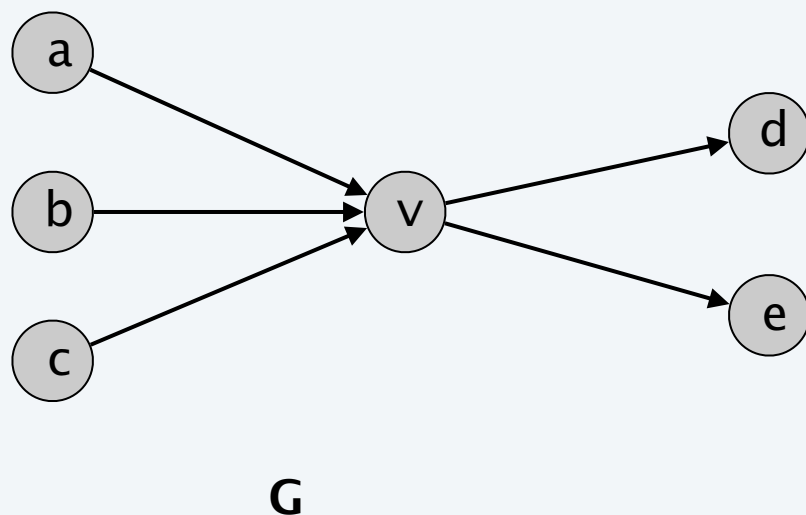
# Directed Hamilton cycle reduces to Hamilton cycle

---

**DIR-HAM-CYCLE:** Given a digraph  $G = (V, E)$ , does there exist a simple directed cycle  $\Gamma$  that contains every node in  $V$ ?

**Theorem.**  $\text{DIR-HAM-CYCLE} \leq_P \text{HAM-CYCLE}$ .

**Pf.** Given a digraph  $G = (V, E)$ , construct a graph  $G'$  with  $3n$  nodes.



# Directed Hamilton cycle reduces to Hamilton cycle

---

**Lemma.**  $G$  has a directed Hamilton cycle iff  $G'$  has a Hamilton cycle.

**Pf.**  $\Rightarrow$

- Suppose  $G$  has a directed Hamilton cycle  $\Gamma$ .
- Then  $G'$  has an undirected Hamilton cycle (same order).

**Pf.**  $\Leftarrow$

- Suppose  $G'$  has an undirected Hamilton cycle  $\Gamma'$ .
- $\Gamma'$  must visit nodes in  $G'$  using one of following two orders:  
 $\dots, \textit{black}, \textit{white}, \textit{blue}, \textit{black}, \textit{white}, \textit{blue}, \textit{black}, \textit{white}, \textit{blue}, \dots$   
 $\dots, \textit{black}, \textit{blue}, \textit{white}, \textit{black}, \textit{blue}, \textit{white}, \textit{black}, \textit{blue}, \textit{white}, \dots$
- Black nodes in  $\Gamma'$  make up directed Hamilton cycle  $\Gamma$  in  $G$ , or reverse of one. ■

## 3-satisfiability reduces to directed Hamilton cycle

---

**Theorem.**  $3\text{-SAT} \leq_P \text{DIR-HAM-CYCLE}$ .

**Pf.** Given an instance  $\Phi$  of 3-SAT, we construct an instance of DIR-HAM-CYCLE that has a Hamilton cycle iff  $\Phi$  is satisfiable.

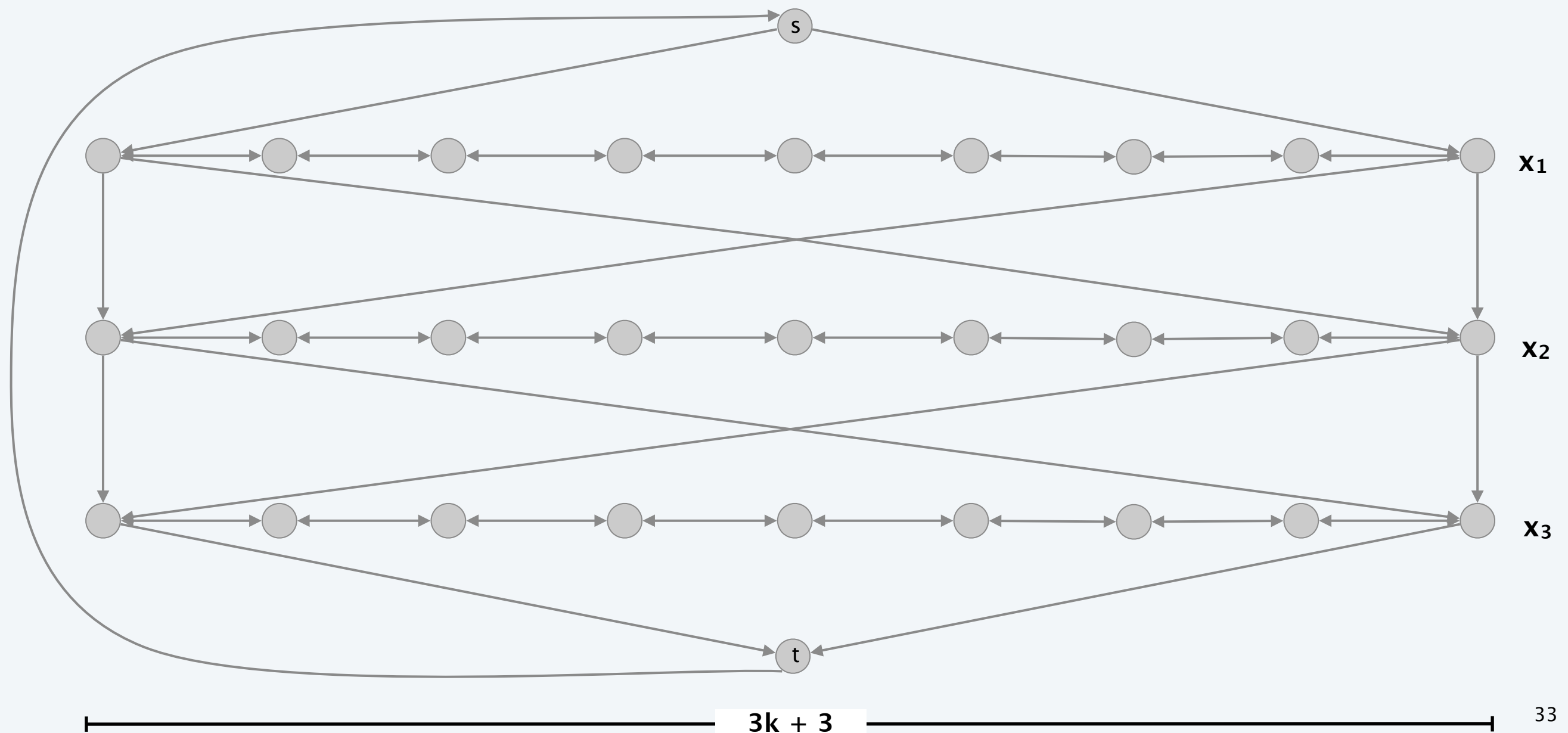
**Construction overview.** Let  $n$  denote the number of variables in  $\Phi$ . We will create graph that has  $2^n$  Hamilton cycles which correspond in a natural way to  $2^n$  possible truth assignments.



# 3-satisfiability reduces to directed Hamilton cycle

**Construction.** Given 3-SAT instance  $\Phi$  with  $n$  variables  $x_i$  and  $k$  clauses.

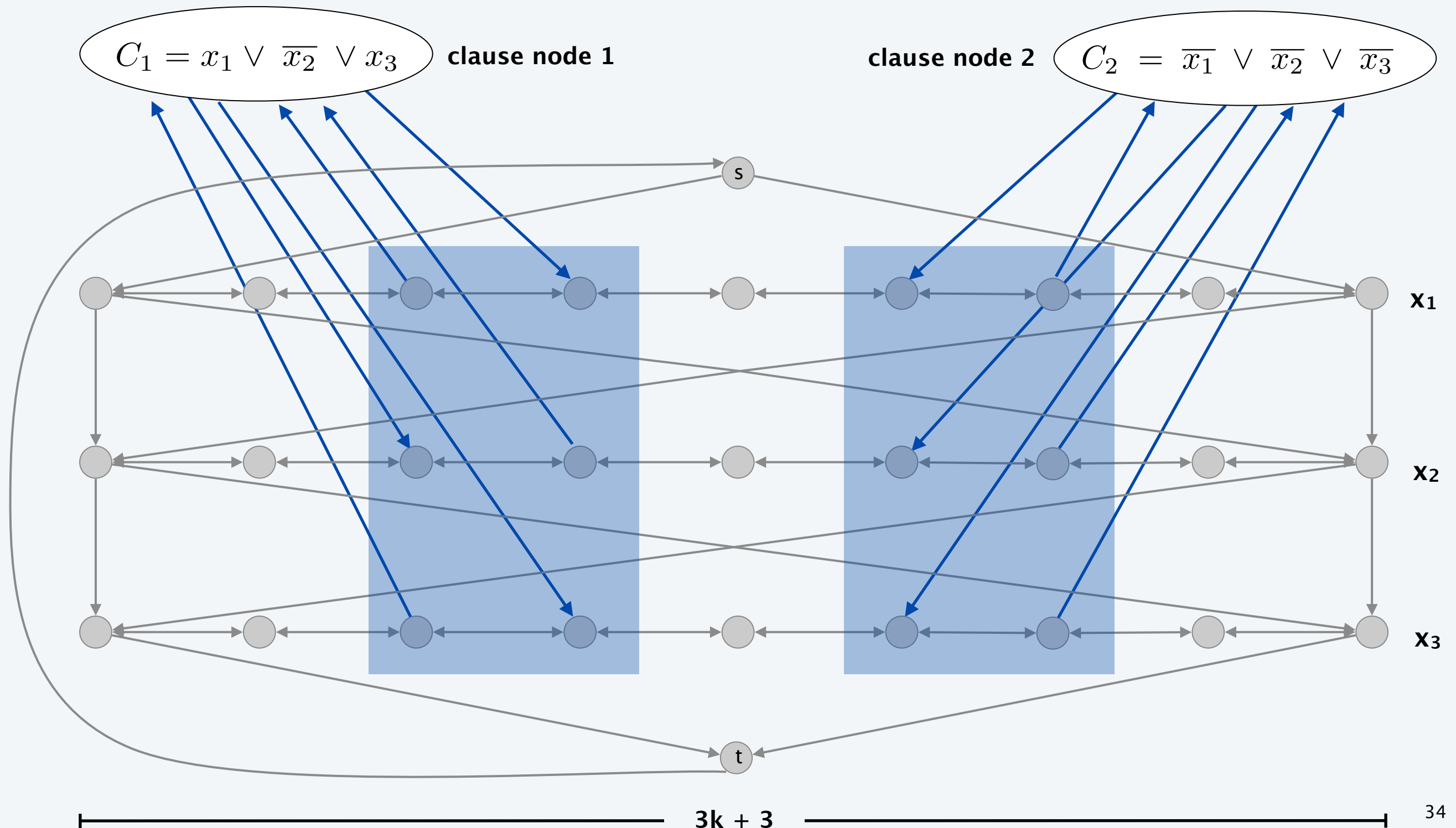
- Construct  $G$  to have  $2^n$  Hamilton cycles.
- Intuition: traverse path  $i$  from left to right  $\Leftrightarrow$  set variable  $x_i = \text{true}$ .



# 3-satisfiability reduces to directed Hamilton cycle

**Construction.** Given 3-SAT instance  $\Phi$  with  $n$  variables  $x_i$  and  $k$  clauses.

- For each clause, add a node and 6 edges.



## 3-satisfiability reduces to directed Hamilton cycle

---

**Lemma.**  $\Phi$  is satisfiable iff  $G$  has a Hamilton cycle.

**Pf.**  $\Rightarrow$

- Suppose 3-SAT instance has satisfying assignment  $x^*$ .
- Then, define Hamilton cycle in  $G$  as follows:
  - if  $x^*_i = \text{true}$ , traverse row  $i$  from left to right
  - if  $x^*_i = \text{false}$ , traverse row  $i$  from right to left
  - for each clause  $C_j$ , there will be at least one row  $i$  in which we are going in “correct” direction to splice clause node  $C_j$  into cycle (and we splice in  $C_j$  exactly once)

## 3-satisfiability reduces to directed Hamilton cycle

---

**Lemma.**  $\Phi$  is satisfiable iff  $G$  has a Hamilton cycle.

**Pf.**  $\Leftarrow$

- Suppose  $G$  has a Hamilton cycle  $\Gamma$ .
- If  $\Gamma$  enters clause node  $C_j$ , it must depart on mate edge.
  - nodes immediately before and after  $C_j$  are connected by an edge  $e \in E$
  - removing  $C_j$  from cycle, and replacing it with edge  $e$  yields Hamilton cycle on  $G - \{ C_j \}$
- Continuing in this way, we are left with a Hamilton cycle  $\Gamma'$  in  $G - \{ C_1, C_2, \dots, C_k \}$ .
- Set  $x^*_i = \text{true}$  iff  $\Gamma'$  traverses row  $i$  left to right.
- Since  $\Gamma$  visits each clause node  $C_j$ , at least one of the paths is traversed in “correct” direction, and each clause is satisfied. ■

## 3-satisfiability reduces to longest path

---

**LONGEST-PATH.** Given a directed graph  $G = (V, E)$ , does there exist a simple path consisting of **at least**  $k$  edges?

**Theorem.**  $3\text{-SAT} \leq_P \text{LONGEST-PATH}$ .

**Pf 1.** Redo proof for DIR-HAM-CYCLE, ignoring back-edge from  $t$  to  $s$ .

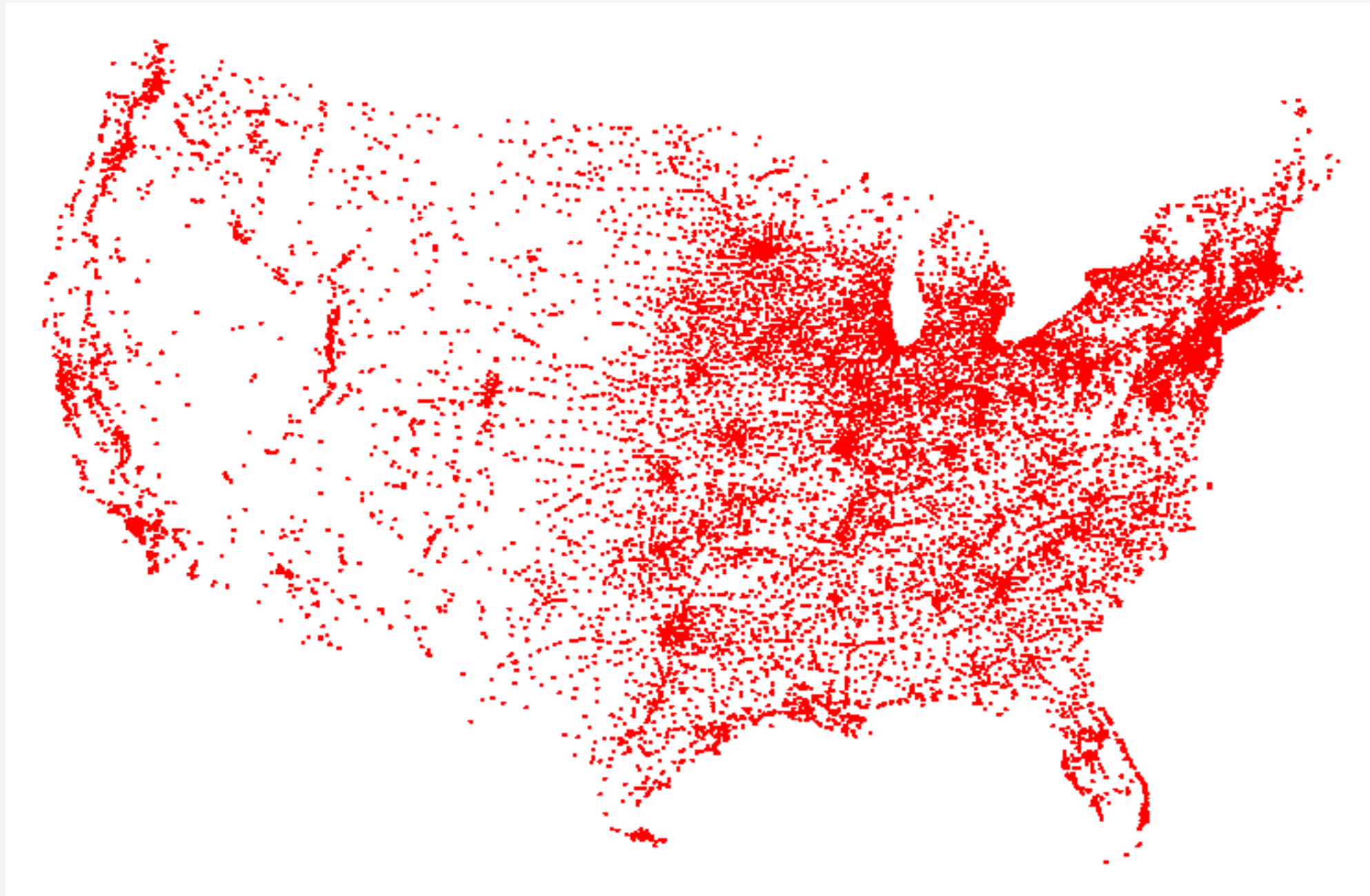
**Pf 2.** Show  $\text{HAM-CYCLE} \leq_P \text{LONGEST-PATH}$ .

# Traveling salesperson problem

---

**TSP.** Given a set of  $n$  cities and a pairwise distance function  $d(u, v)$ ,  
is there a tour of length  $\leq D$ ?

can view as a complete graph

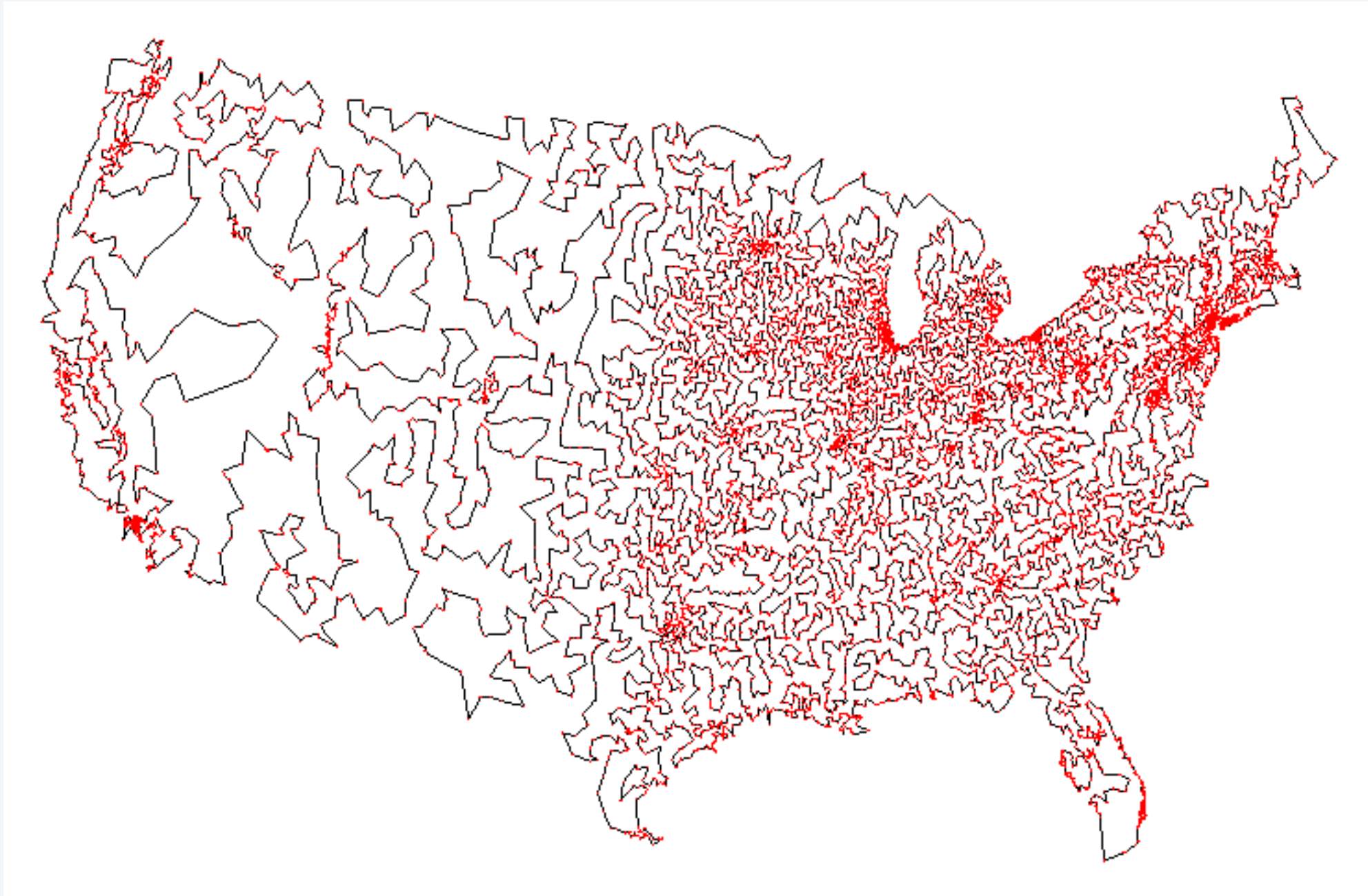


13,509 cities in the United States  
<http://www.math.uwaterloo.ca/tsp>

# Traveling salesperson problem

---

**TSP.** Given a set of  $n$  cities and a pairwise distance function  $d(u, v)$ , is there a tour of length  $\leq D$ ?



optimal TSP tour

<http://www.math.uwaterloo.ca/tsp>



# Traveling salesperson problem

---

**TSP.** Given a set of  $n$  cities and a pairwise distance function  $d(u, v)$ , is there a tour of length  $\leq D$ ?



11,849 holes to drill in a programmed logic array

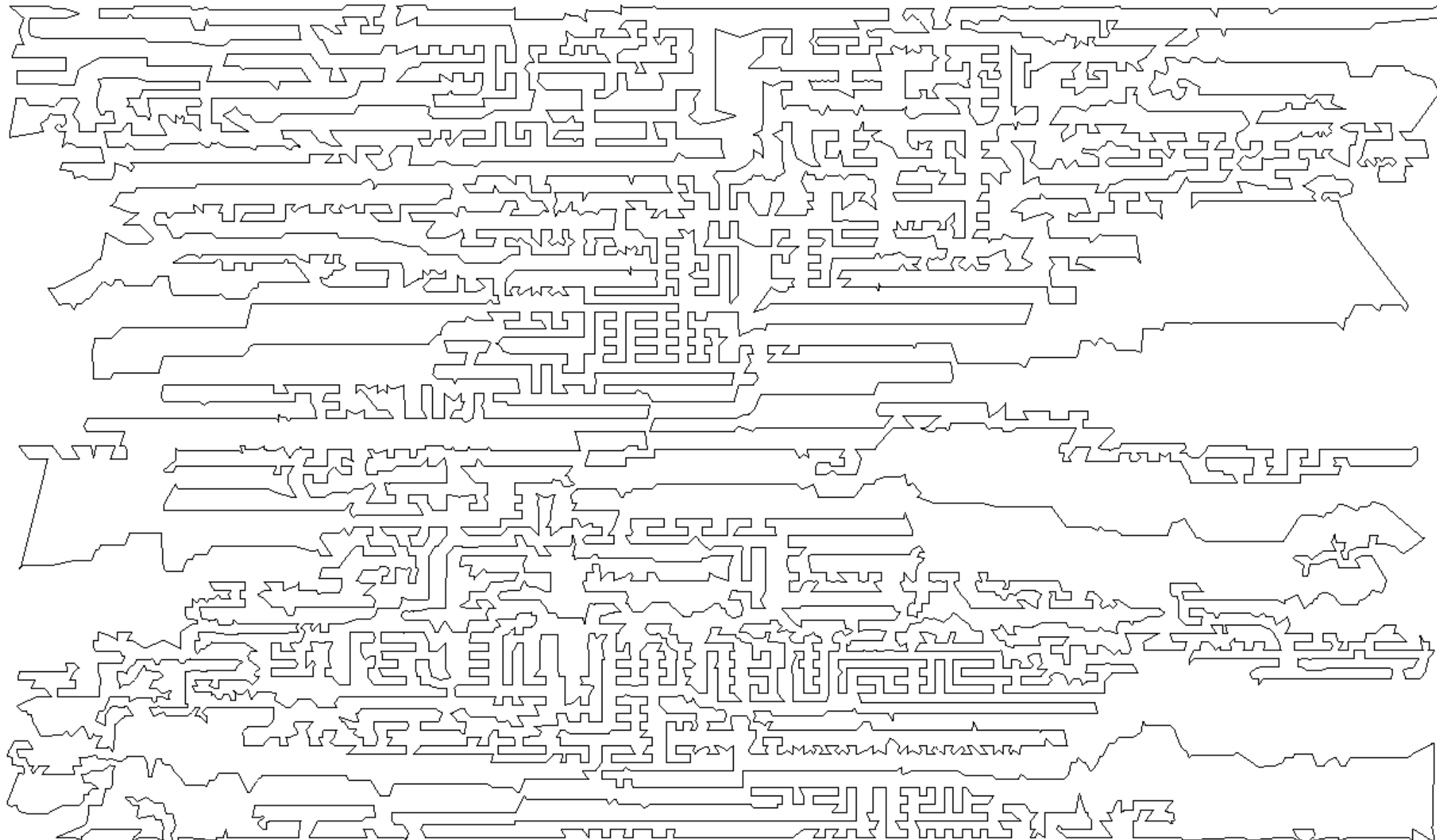
<http://www.math.uwaterloo.ca/tsp>



# Traveling salesperson problem

---

**TSP.** Given a set of  $n$  cities and a pairwise distance function  $d(u, v)$ , is there a tour of length  $\leq D$ ?



**optimal TSP tour**

<http://www.math.uwaterloo.ca/tsp>

# Hamilton cycle reduces to traveling salesperson problem

---

**TSP.** Given a set of  $n$  cities and a pairwise distance function  $d(u, v)$ , is there a tour of length  $\leq D$ ?

**HAM-CYCLE.** Given an undirected graph  $G = (V, E)$ , does there exist a simple cycle  $\Gamma$  that contains every node in  $V$ ?

**Theorem.**  $\text{HAM-CYCLE} \leq_P \text{TSP}$ .

**Pf.**

- Given an instance  $G = (V, E)$  of HAM-CYCLE, create  $n = |V|$  cities with distance function

$$d(u, v) = \begin{cases} 1 & \text{if } (u, v) \in E \\ 2 & \text{if } (u, v) \notin E \end{cases}$$

- TSP instance has tour of length  $\leq n$  iff  $G$  has a Hamilton cycle. ■

**Remark.** TSP instance satisfies triangle inequality:  $d(u, w) \leq d(u, v) + d(v, w)$ .

# Polynomial-time reductions

---

constraint satisfaction

3-SAT

*3-SAT poly-time reduces to INDEPENDENT-SET*

INDEPENDENT-SET

DIR-HAM-CYCLE

GRAPH-3-COLOR

SUBSET-SUM

VERTEX-COVER

HAM-CYCLE

PLANAR-3-COLOR

SCHEDULING

SET-COVER

TSP

packing and covering

sequencing

partitioning

numerical



## 8. INTRACTABILITY I

---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ ***partitioning problems***
- ▶ *graph coloring*
- ▶ *numerical problems*

# 3-dimensional matching

---

**3D-MATCHING.** Given  $n$  instructors,  $n$  courses, and  $n$  times, and a list of the possible courses and times each instructor is willing to teach, is it possible to make an assignment so that all courses are taught at different times?

instructor	course	time
Wayne	COS 226	TTh 11–12:20
Wayne	COS 423	MW 11–12:20
Wayne	COS 423	TTh 11–12:20
Tardos	COS 423	TTh 3–4:20
Tardos	COS 523	TTh 3–4:20
Kleinberg	COS 226	TTh 3–4:20
Kleinberg	COS 226	MW 11–12:20
Kleinberg	COS 423	MW 11–12:20

## 3-dimensional matching

---

**3D-MATCHING.** Given 3 disjoint sets  $X$ ,  $Y$ , and  $Z$ , each of size  $n$  and a set  $T \subseteq X \times Y \times Z$  of triples, does there exist a set of  $n$  triples in  $T$  such that each element of  $X \cup Y \cup Z$  is in exactly one of these triples?

$$X = \{ x_1, x_2, x_3 \}, \quad Y = \{ y_1, y_2, y_3 \}, \quad Z = \{ z_1, z_2, z_3 \}$$

$$T_1 = \{ x_1, y_1, z_2 \}, \quad T_2 = \{ x_1, y_2, z_1 \}, \quad T_3 = \{ x_1, y_2, z_2 \}$$

$$T_4 = \{ x_2, y_2, z_3 \}, \quad T_5 = \{ x_2, y_3, z_3 \},$$

$$T_7 = \{ x_3, y_1, z_3 \}, \quad T_8 = \{ x_3, y_1, z_1 \}, \quad T_9 = \{ x_3, y_2, z_1 \}$$

an instance of 3d-matching (with  $n = 3$ )

**Remark.** Generalization of bipartite matching.

## 3-dimensional matching

---

**3D-MATCHING.** Given 3 disjoint sets  $X$ ,  $Y$ , and  $Z$ , each of size  $n$  and a set  $T \subseteq X \times Y \times Z$  of triples, does there exist a set of  $n$  triples in  $T$  such that each element of  $X \cup Y \cup Z$  is in exactly one of these triples?

**Theorem.**  $3\text{-SAT} \leq_P 3\text{D-MATCHING}$ .

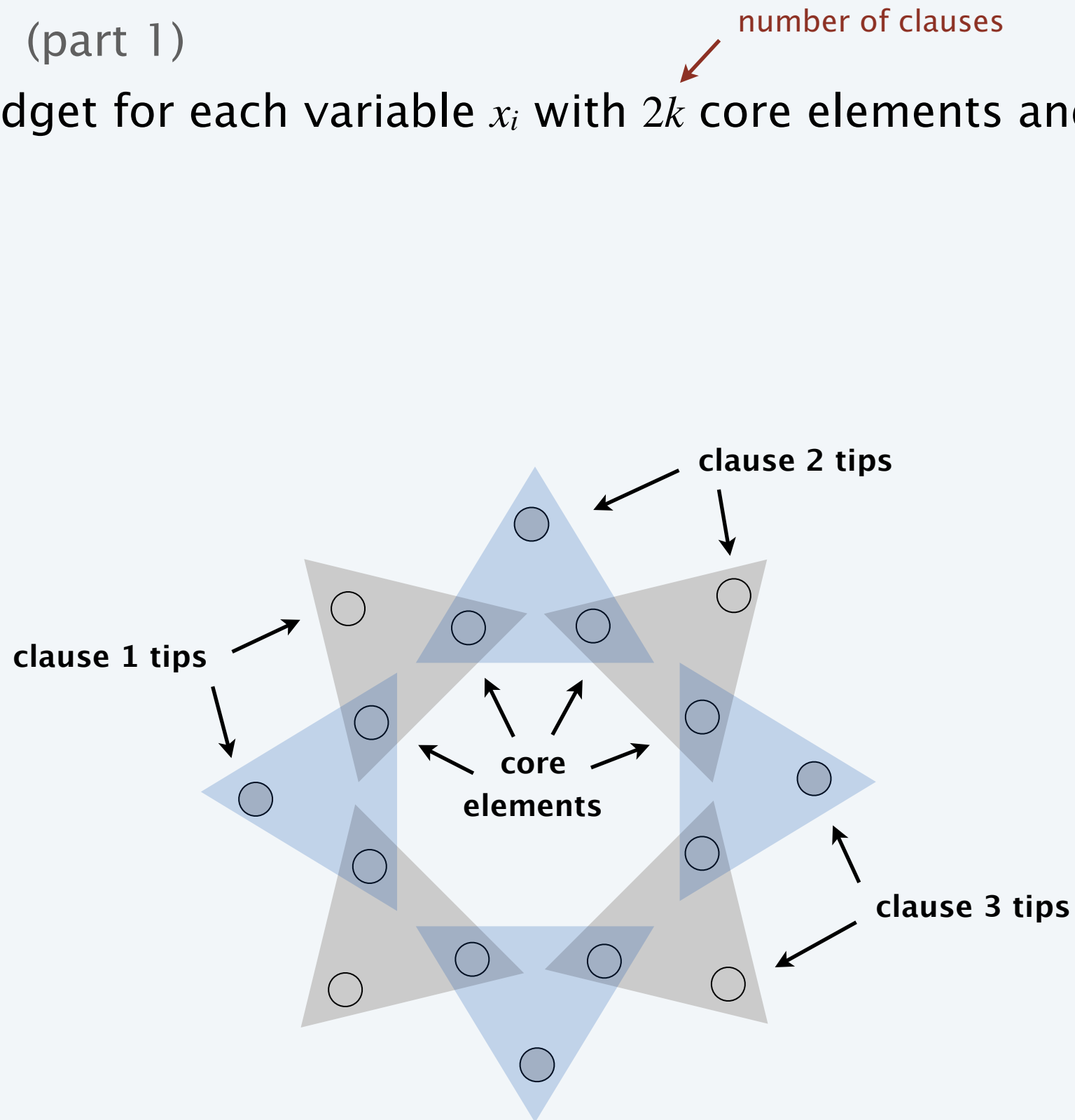
**Pf.** Given an instance  $\Phi$  of 3-SAT, we construct an instance of 3D-MATCHING that has a perfect matching iff  $\Phi$  is satisfiable.

# 3-satisfiability reduces to 3-dimensional matching

---

## Construction. (part 1)

- Create gadget for each variable  $x_i$  with  $2k$  core elements and  $2k$  tip ones.



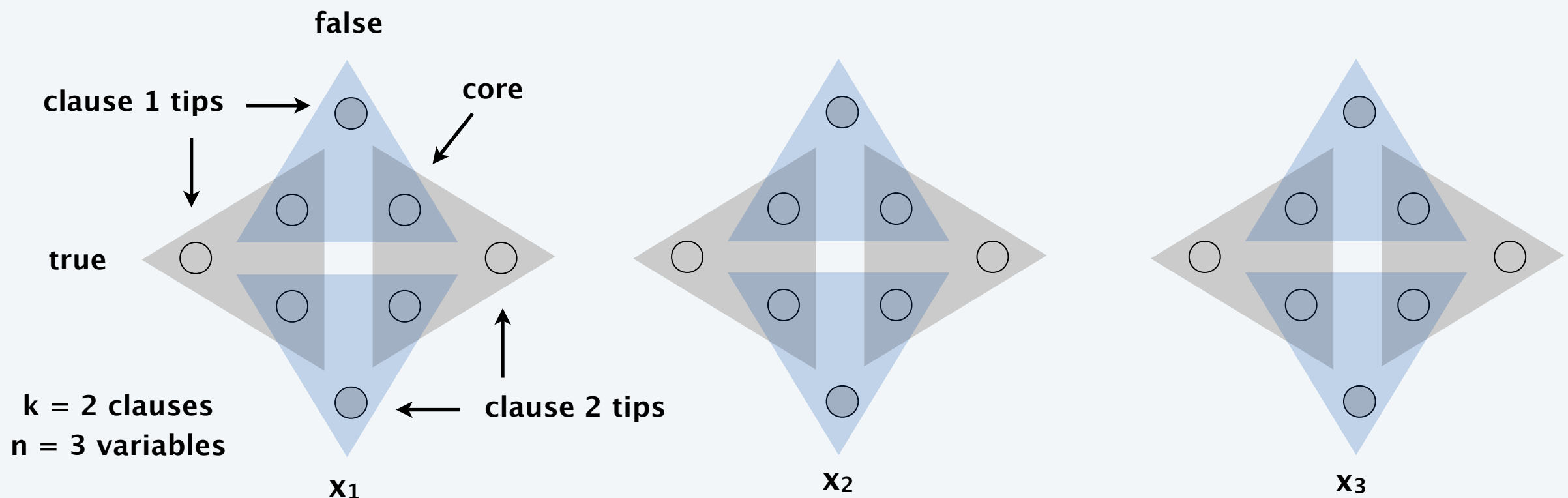
a gadget for variable  $x_i$  ( $k = 4$ )



# 3-satisfiability reduces to 3-dimensional matching

## Construction. (part 1)

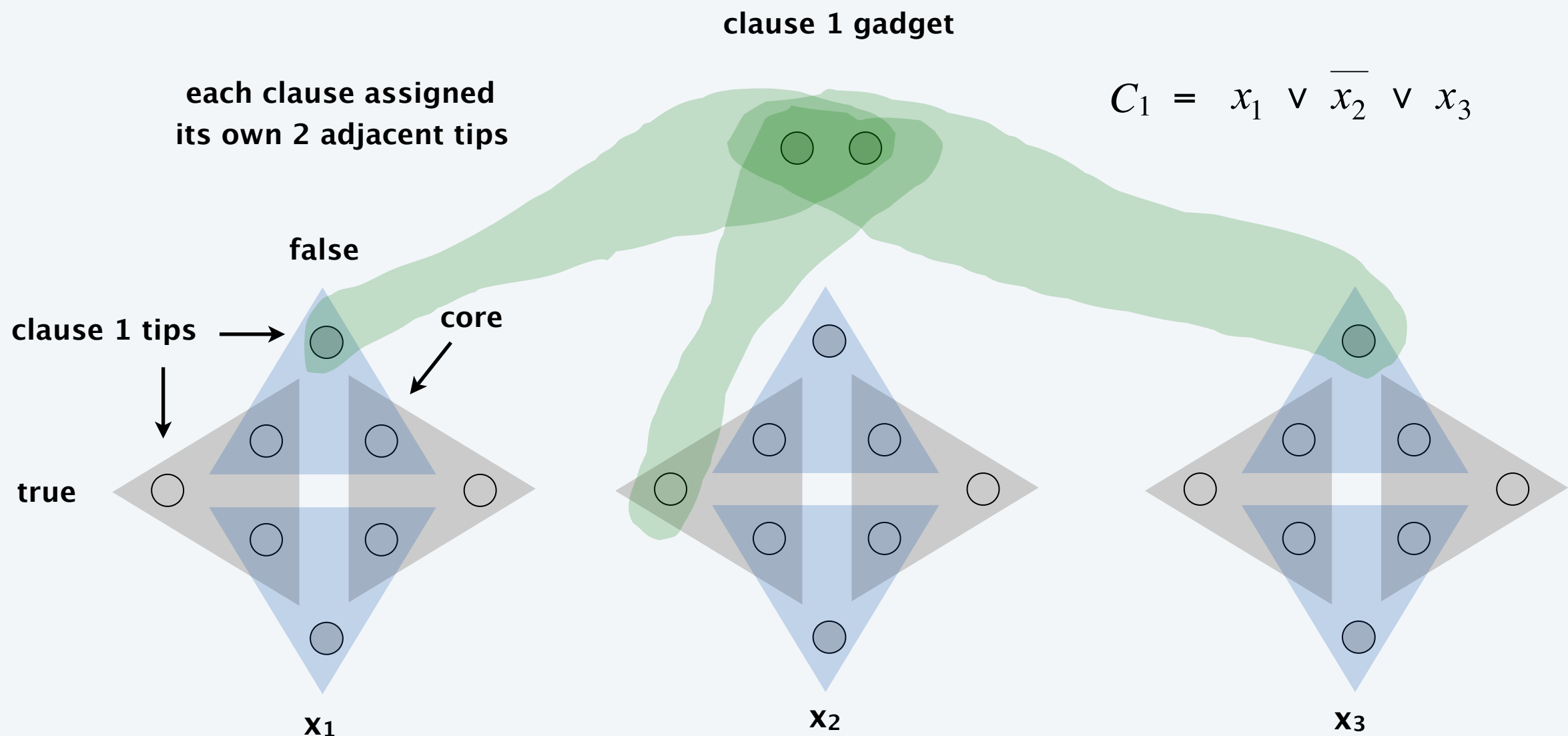
- Create gadget for each variable  $x_i$  with  $2k$  core elements and  $2k$  tip ones.
- No other triples will use core elements.
- In gadget for  $x_i$ , any perfect matching must use either all gray triples (corresponding to  $x_i = \text{true}$ ) or all blue ones (corresponding to  $x_i = \text{false}$ ).



# 3-satisfiability reduces to 3-dimensional matching

## Construction. (part 2)

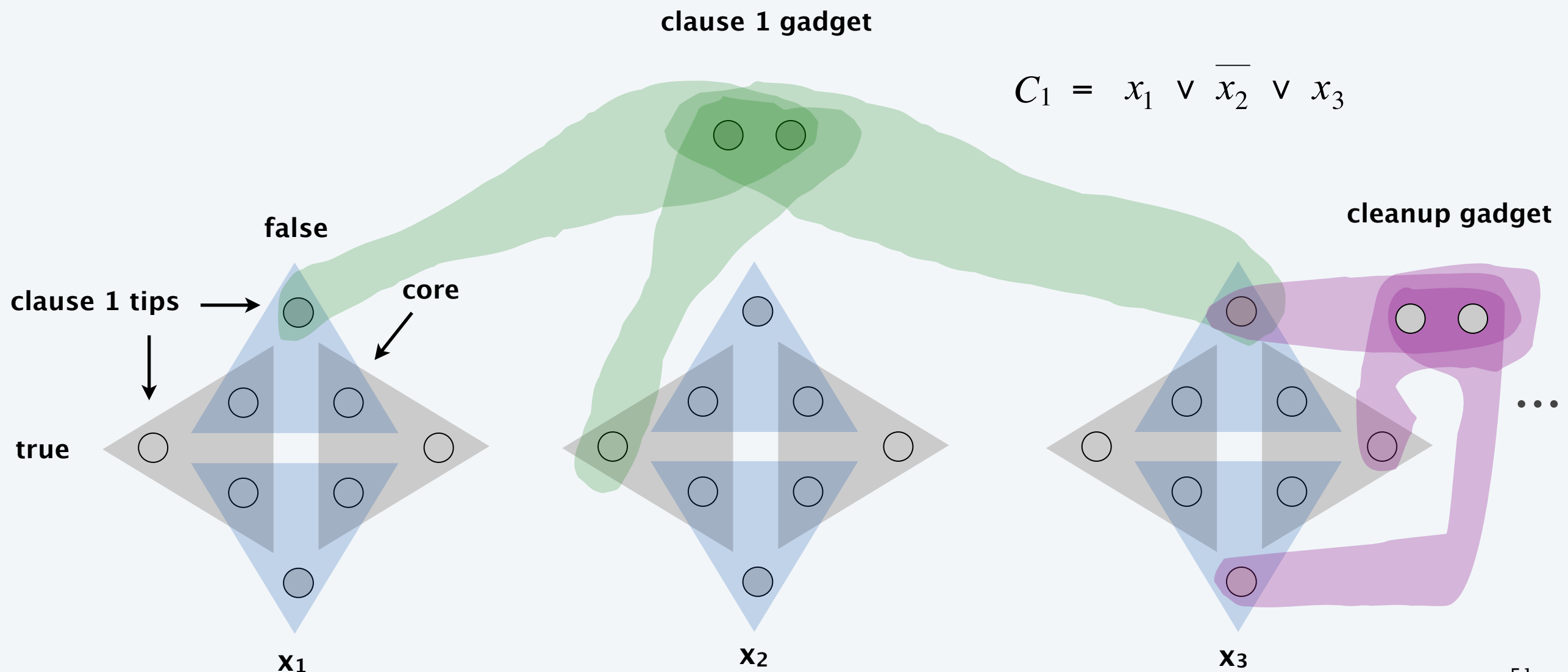
- Create gadget for each clause  $C_j$  with two elements and three triples.
- Exactly one of these triples will be used in any 3d-matching.
- Ensures any perfect matching uses either (i) grey core of  $x_1$  or (ii) blue core of  $x_2$  or (iii) grey core of  $x_3$ .



# 3-satisfiability reduces to 3-dimensional matching

## Construction. (part 3)

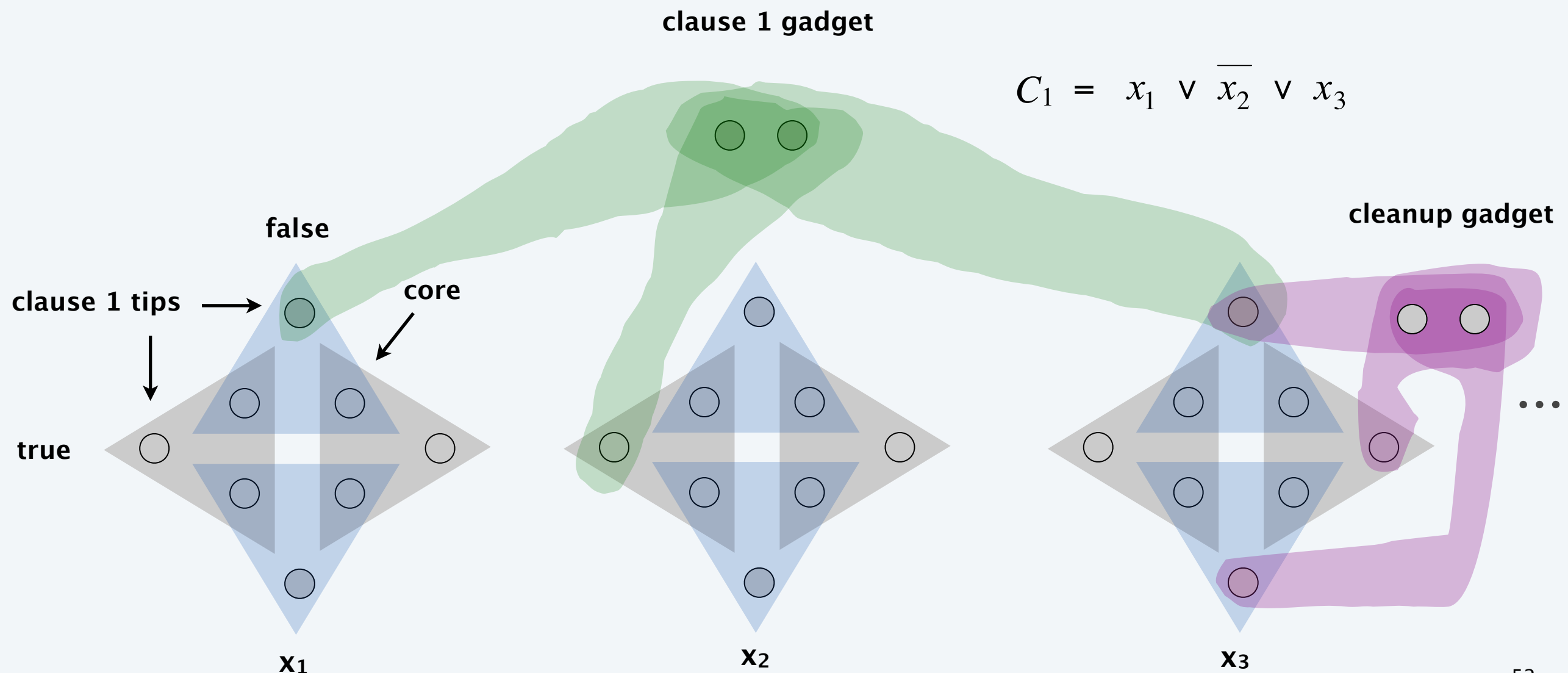
- There are  $2nk$  tips:  $nk$  covered by blue/gray triples;  $k$  by clause triples.
- To cover remaining  $(n-1)k$  tips, create  $(n-1)k$  cleanup gadgets: same as clause gadget but with  $2nk$  triples, connected to every tip.



# 3-satisfiability reduces to 3-dimensional matching

**Lemma.** Instance  $(X, Y, Z)$  has a perfect matching iff  $\Phi$  is satisfiable.

**Q.** What are  $X$ ,  $Y$ , and  $Z$ ?

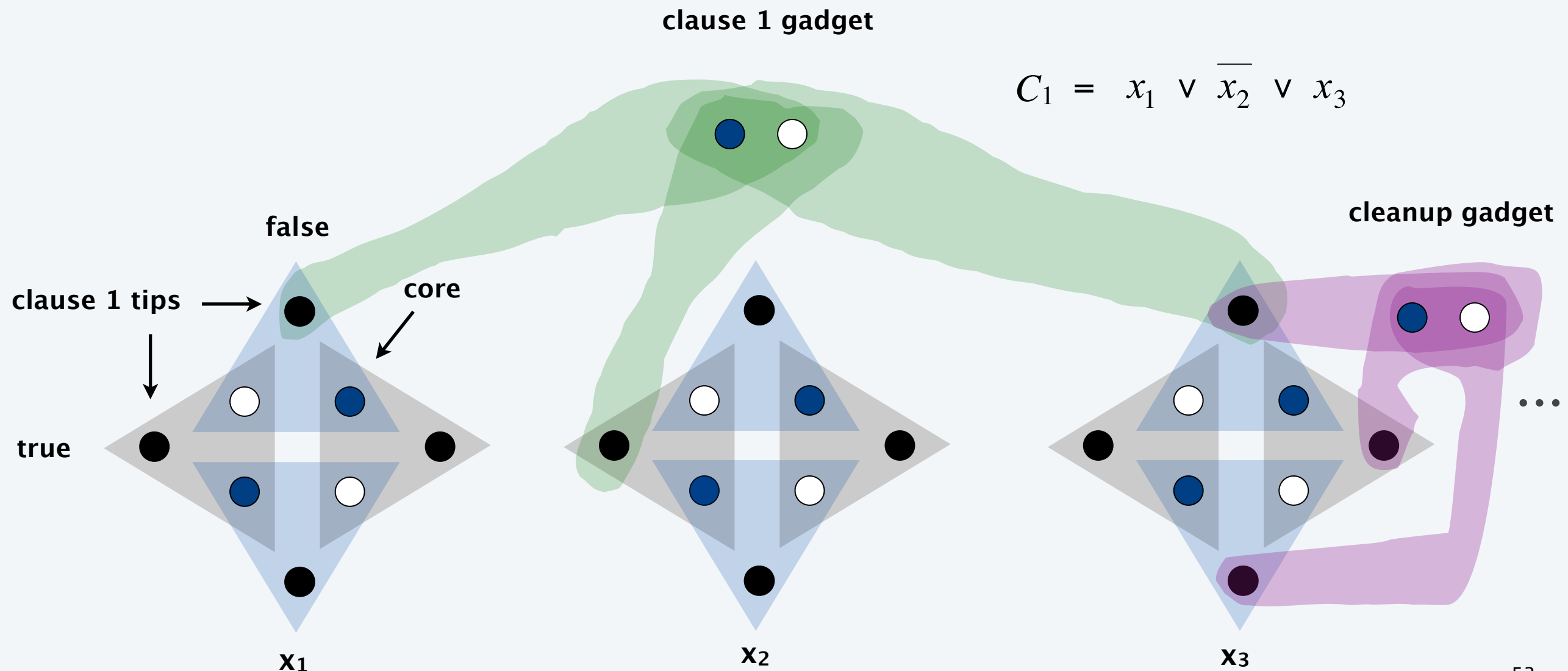


## 3-satisfiability reduces to 3-dimensional matching

**Lemma.** Instance  $(X, Y, Z)$  has a perfect matching iff  $\Phi$  is satisfiable.

Q. What are  $X$ ,  $Y$ , and  $Z$ ?

**A.**  $X = \text{black}$ ,  $Y = \text{white}$ , and  $Z = \text{blue}$ .

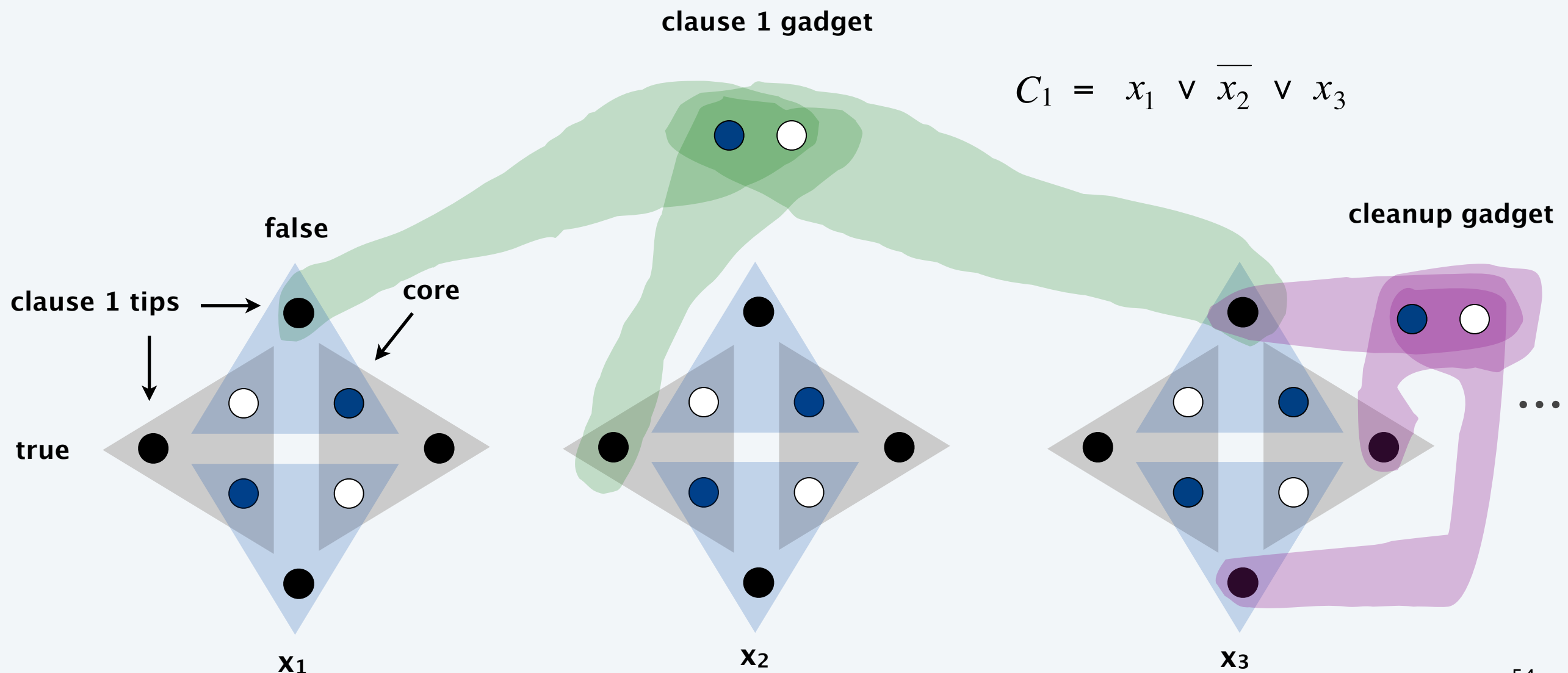


# 3-satisfiability reduces to 3-dimensional matching

**Lemma.** Instance  $(X, Y, Z)$  has a perfect matching iff  $\Phi$  is satisfiable.

**Pf.  $\Rightarrow$**  If 3d-matching, then assign  $x_i$  according to gadget  $x_i$ .

**Pf.  $\Leftarrow$**  If  $\Phi$  is satisfiable, use any true literal in  $C_j$  to select gadget  $C_j$  triple. ■





## 8. INTRACTABILITY I

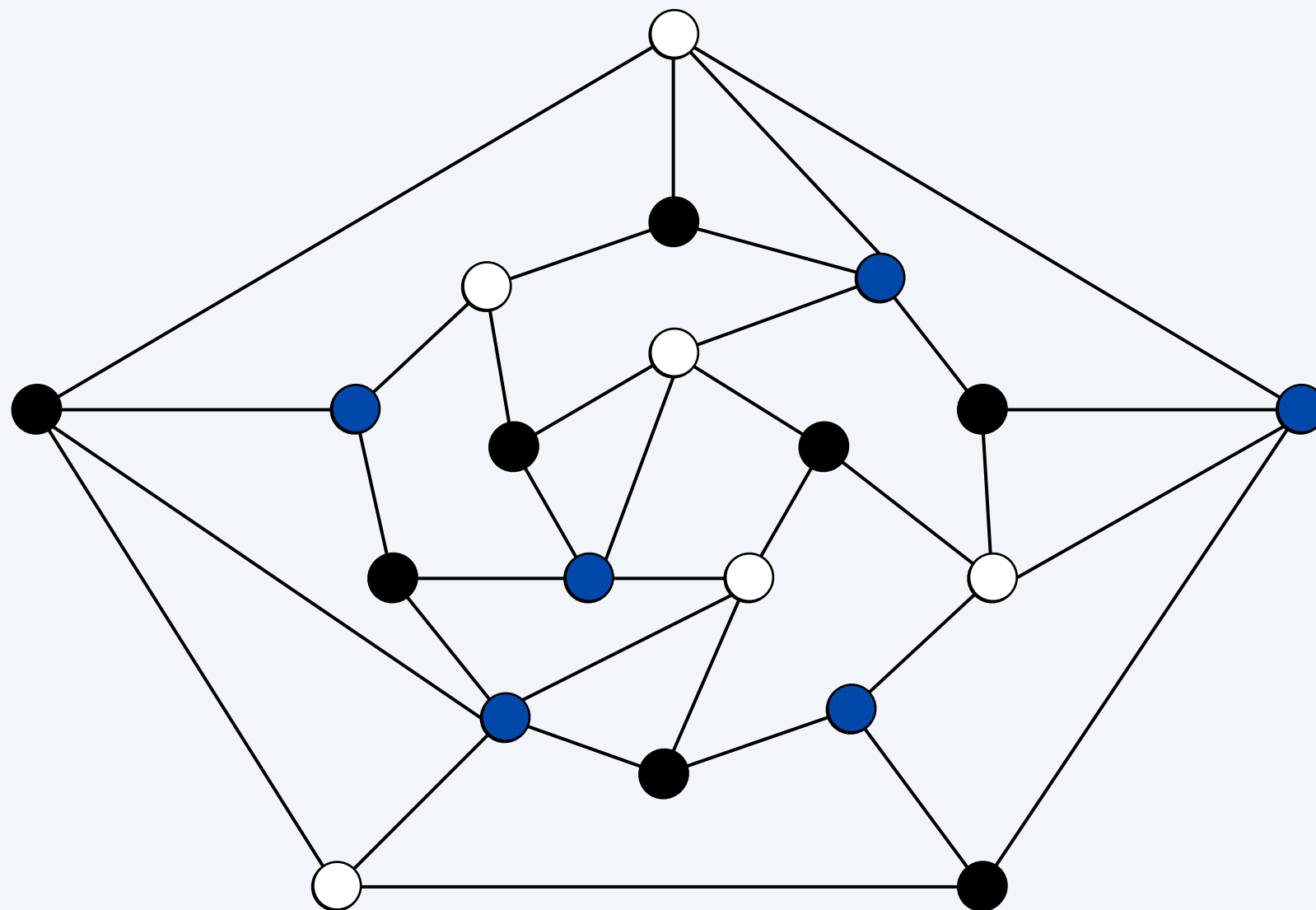
---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ ***graph coloring***
- ▶ *numerical problems*

# 3-colorability

---

**3-COLOR.** Given an undirected graph  $G$ , can the nodes be colored black, white, and blue so that no adjacent nodes have the same color?



yes instance



# Application: register allocation

---

**Register allocation.** Assign program variables to machine registers so that no more than  $k$  registers are used and no two program variables that are needed at the same time are assigned to the same register.

**Interference graph.** Nodes are program variables; edge between  $u$  and  $v$  if there exists an operation where both  $u$  and  $v$  are “live” at the same time.

**Observation.** [Chaitin 1982] Can solve register allocation problem iff interference graph is  $k$ -colorable.

**Fact.**  $3\text{-COLOR} \leq_P K\text{-REGISTER-ALLOCATION}$  for any constant  $k \geq 3$ .

REGISTER ALLOCATION & SPILLING VIA GRAPH COLORING

G. J. Chaitin  
IBM Research  
P.O.Box 218, Yorktown Heights, NY 10598

## 3-satisfiability reduces to 3-colorability

---

**Theorem.**  $3\text{-SAT} \leq_p 3\text{-COLOR}$ .

**Pf.** Given 3-SAT instance  $\Phi$ , we construct an instance of 3-COLOR that is 3-colorable iff  $\Phi$  is satisfiable.

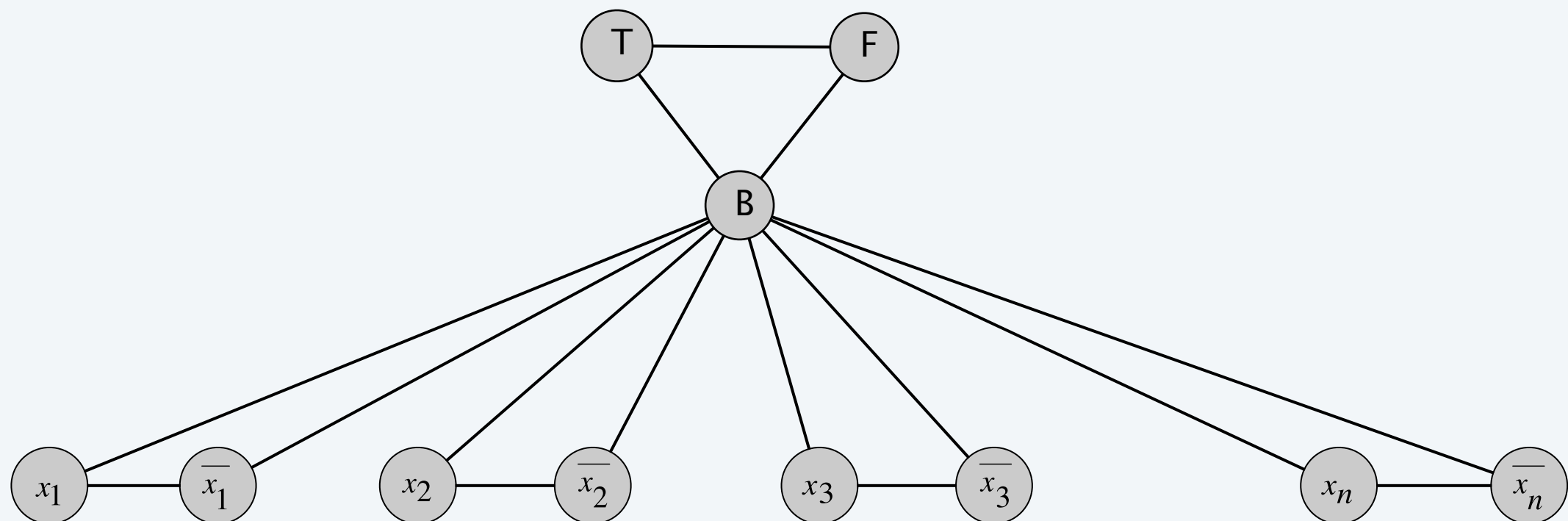
# 3-satisfiability reduces to 3-colorability

---

## Construction.

- (i) Create a graph  $G$  with a node for each literal.
- (ii) Connect each literal to its negation.
- (iii) Create 3 new nodes  $T$ ,  $F$ , and  $B$ ; connect them in a triangle.
- (iv) Connect each literal to  $B$ .
- (v) For each clause  $C_j$ , add a gadget of 6 nodes and 13 edges.

↑  
to be described later



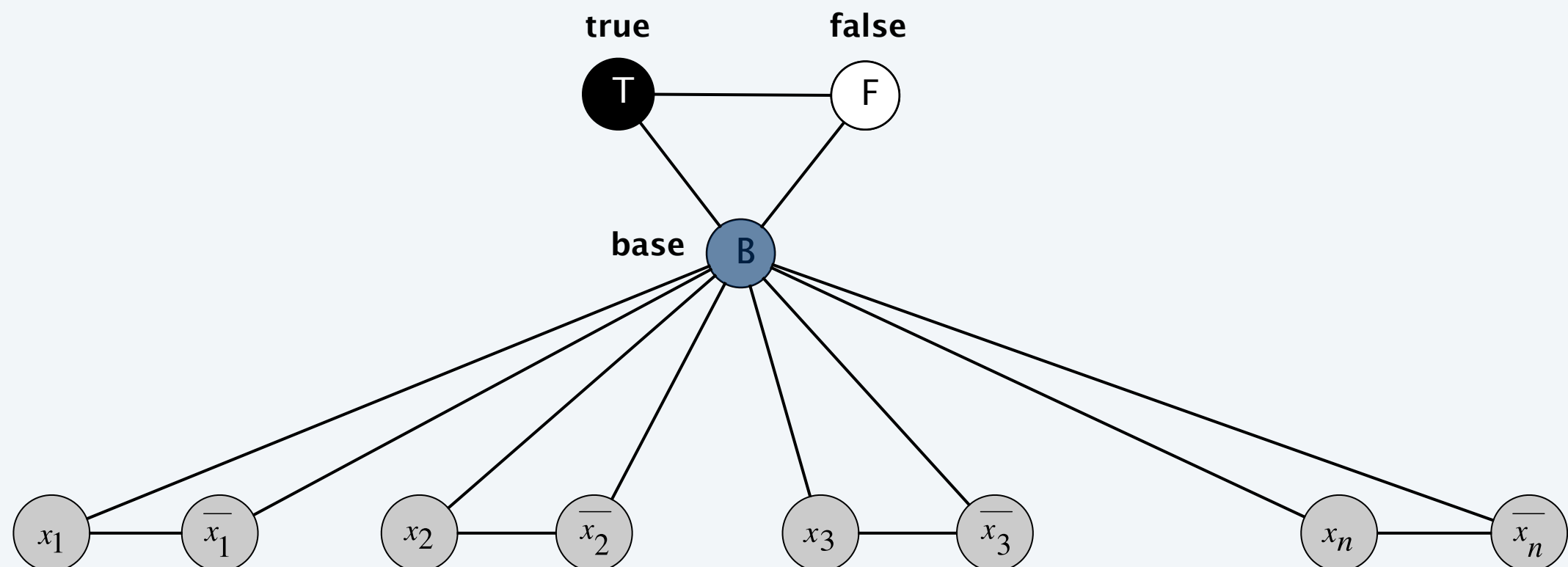
# 3-satisfiability reduces to 3-colorability

---

**Lemma.** Graph  $G$  is 3-colorable iff  $\Phi$  is satisfiable.

**Pf.**  $\Rightarrow$  Suppose graph  $G$  is 3-colorable.

- WLOG, assume that node  $T$  is colored *black*,  $F$  is *white*, and  $B$  is *blue*.
- Consider assignment that sets all *black* literals to *true* (and *white* to *false*).
- (iv) ensures each literal is colored either *black* or *white*.
- (ii) ensures that each literal is *white* if its negation is *black* (and vice versa).

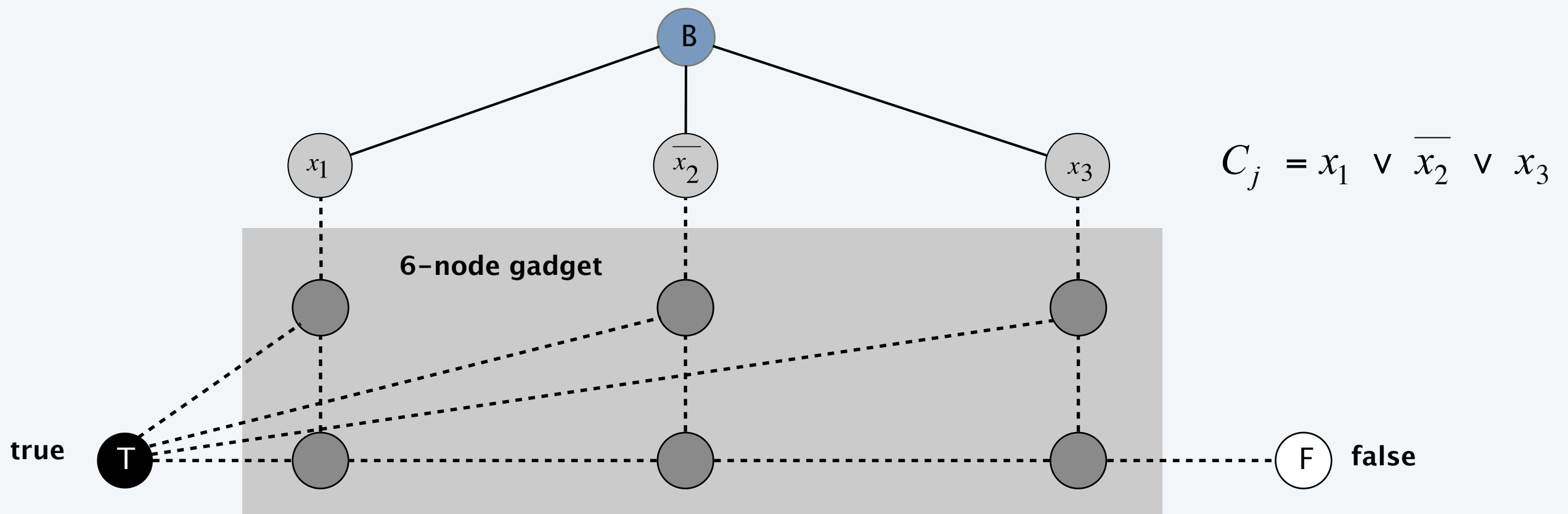


# 3-satisfiability reduces to 3-colorability

**Lemma.** Graph  $G$  is 3-colorable iff  $\Phi$  is satisfiable.

**Pf.**  $\Rightarrow$  Suppose graph  $G$  is 3-colorable.

- WLOG, assume that node  $T$  is colored *black*,  $F$  is *white*, and  $B$  is *blue*.
- Consider assignment that sets all *black* literals to *true* (and *white* to *false*).
- (iv) ensures each literal is colored either *black* or *white*.
- (ii) ensures that each literal is *white* if its negation is *black* (and vice versa).
- (v) ensures at least one literal in each clause is *black*.

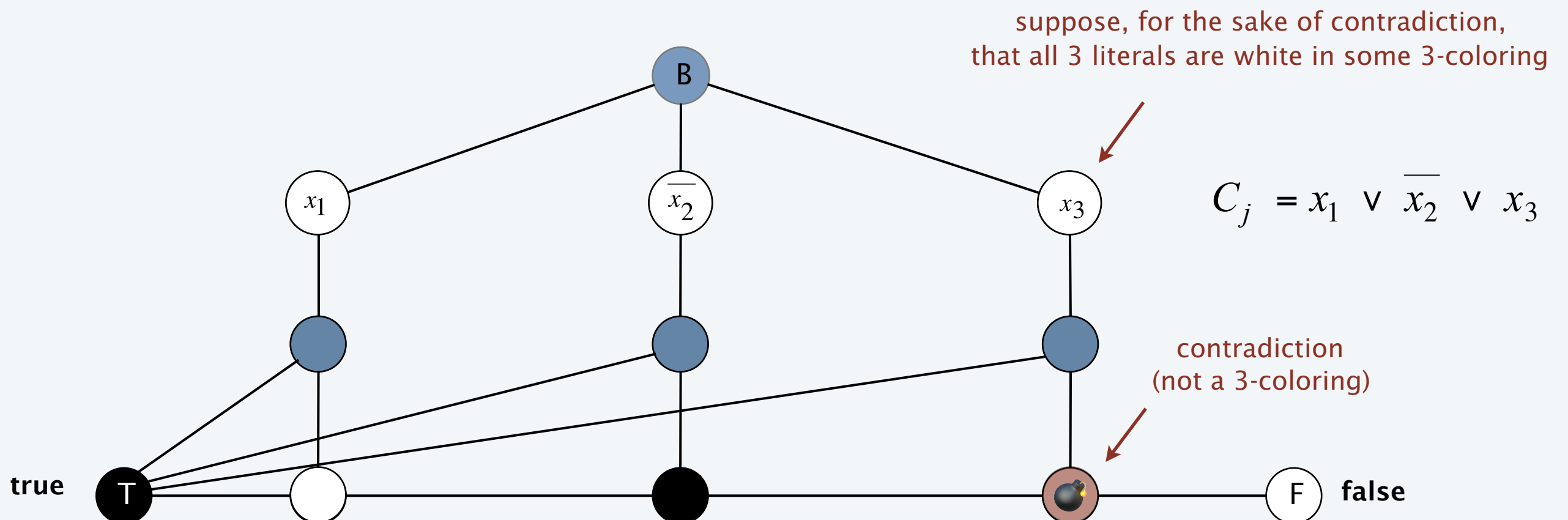


# 3-satisfiability reduces to 3-colorability

**Lemma.** Graph  $G$  is 3-colorable iff  $\Phi$  is satisfiable.

Pf.  $\Rightarrow$  Suppose graph  $G$  is 3-colorable.

- WLOG, assume that node  $T$  is colored *black*,  $F$  is *white*, and  $B$  is *blue*.
- Consider assignment that sets all *black* literals to *true* (and *white* to *false*).
- (iv) ensures each literal is colored either *black* or *white*.
- (ii) ensures that each literal is *white* if its negation is *black* (and vice versa).
- (v) ensures at least one literal in each clause is *black*.

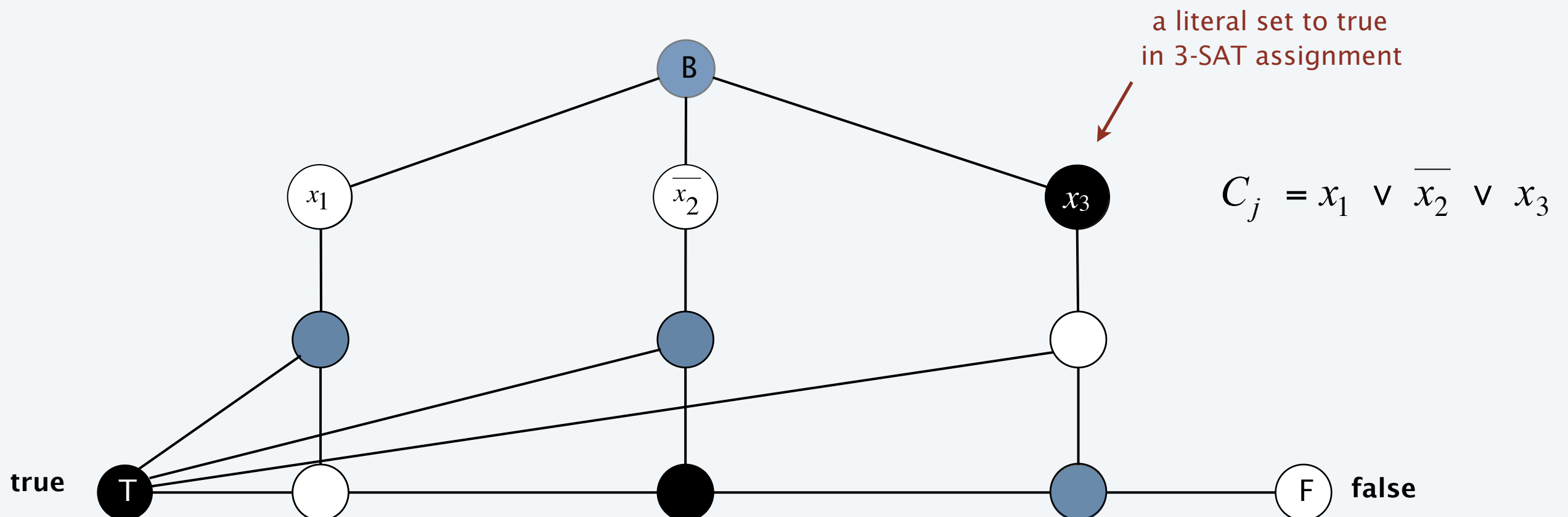


# 3-satisfiability reduces to 3-colorability

**Lemma.** Graph  $G$  is 3-colorable iff  $\Phi$  is satisfiable.

**Pf.**  $\Leftarrow$  Suppose 3-SAT instance  $\Phi$  is satisfiable.

- Color all *true* literals *black* and all *false* literals *white*.
- Pick one *true* literal; color node below that node *white*, and node below that *blue*.
- Color remaining middle row nodes *blue*.
- Color remaining bottom nodes *black* or *white*, as forced. ■



# Polynomial-time reductions

---

constraint satisfaction

3-SAT

*3-SAT poly-time reduces  
to INDEPENDENT-SET*

INDEPENDENT-SET

DIR-HAM-CYCLE

GRAPH-3-COLOR

SUBSET-SUM

VERTEX-COVER

HAM-CYCLE

PLANAR-3-COLOR

SCHEDULING

SET-COVER

TSP

packing and covering

sequencing

partitioning

numerical





## 8. INTRACTABILITY I

---

- ▶ *poly-time reductions*
- ▶ *packing and covering problems*
- ▶ *constraint satisfaction problems*
- ▶ *sequencing problems*
- ▶ *partitioning problems*
- ▶ *graph coloring*
- ▶ ***numerical problems***

# Subset sum

---

**SUBSET-SUM.** Given natural numbers  $w_1, \dots, w_n$  and an integer  $W$ , is there a subset that adds up to exactly  $W$ ?

**Ex.**  $\{ 1, 4, 16, 64, 256, 1040, 1041, 1093, 1284, 1344 \}$ ,  $W = 3754$ .

**Yes.**  $1 + 16 + 64 + 256 + 1040 + 1093 + 1284 = 3754$ .

**Remark.** With arithmetic problems, input integers are encoded in binary. Poly-time reduction must be polynomial in **binary** encoding.

# Subset sum

---

**Theorem.**  $3\text{-SAT} \leq_P \text{SUBSET-SUM}$ .

**Pf.** Given an instance  $\Phi$  of 3-SAT, we construct an instance of SUBSET-SUM that has solution iff  $\Phi$  is satisfiable.

# 3-satisfiability reduces to subset sum

**Construction.** Given 3-SAT instance  $\Phi$  with  $n$  variables and  $k$  clauses, form  $2n + 2k$  decimal integers, each of  $n + k$  digits:

- Include one digit for each variable  $x_i$  and for each clause  $C_j$ .
- Include two numbers for each variable  $x_i$ .
- Include two numbers for each clause  $C_j$ .
- Sum of each  $x_i$  digit is 1;  
sum of each  $C_j$  digit is 4.

**Key property.** No carries possible  $\Rightarrow$  each digit yields one equation.

$$\begin{aligned}
 C_1 &= \neg x_1 \vee x_2 \vee x_3 \\
 C_2 &= x_1 \vee \neg x_2 \vee x_3 \\
 C_3 &= \neg x_1 \vee \neg x_2 \vee \neg x_3
 \end{aligned}$$

3-SAT instance

dummies to get clause columns to sum to 4

	$x_1$	$x_2$	$x_3$	$C_1$	$C_2$	$C_3$	
$x_1$	1	0	0	0	1	0	100,010
$\neg x_1$	1	0	0	1	0	1	100,101
$x_2$	0	1	0	1	0	0	10,100
$\neg x_2$	0	1	0	0	1	1	10,011
$x_3$	0	0	1	1	1	0	1,110
$\neg x_3$	0	0	1	0	0	1	1,001
<div> </div>	0	0	0	1	0	0	100
	0	0	0	2	0	0	200
	0	0	0	0	1	0	10
	0	0	0	0	2	0	20
	0	0	0	0	0	1	1
	0	0	0	0	0	2	2
$W$	1	1	1	4	4	4	111,444

SUBSET-SUM instance

# 3-satisfiability reduces to subset sum

**Lemma.**  $\Phi$  is satisfiable iff there exists a subset that sums to  $W$ .

**Pf.**  $\Rightarrow$  Suppose  $\Phi$  is satisfiable.

- Choose integers corresponding to each *true* literal.
- Since  $\Phi$  is satisfiable, each  $C_j$  digit sums to at least 1 from  $x_i$  rows.
- Choose dummy integers to make clause digits sum to 4.

$$\begin{aligned} C_1 &= \neg x_1 \vee x_2 \vee x_3 \\ C_2 &= x_1 \vee \neg x_2 \vee x_3 \\ C_3 &= \neg x_1 \vee \neg x_2 \vee \neg x_3 \end{aligned}$$

3-SAT instance

dummies to get clause columns to sum to 4

	$x_1$	$x_2$	$x_3$	$C_1$	$C_2$	$C_3$	
$x_1$	1	0	0	0	1	0	100,010
$\neg x_1$	1	0	0	1	0	1	100,101
$x_2$	0	1	0	1	0	0	10,100
$\neg x_2$	0	1	0	0	1	1	10,011
$x_3$	0	0	1	1	1	0	1,110
$\neg x_3$	0	0	1	0	0	1	1,001
}	0	0	0	1	0	0	100
	0	0	0	2	0	0	200
	0	0	0	0	1	0	10
	0	0	0	0	2	0	20
	0	0	0	0	0	1	1
	0	0	0	0	0	2	2
$W$	1	1	1	4	4	4	111,444

SUBSET-SUM instance

# 3-satisfiability reduces to subset sum

**Lemma.**  $\Phi$  is satisfiable iff there exists a subset that sums to  $W$ .


**Pf.**  $\Leftarrow$  Suppose there is a subset that sums to  $W$ .

- Digit  $x_i$  forces subset to select either row  $x_i$  or  $\neg x_i$  (but not both).
- Digit  $C_j$  forces subset to select at least one literal in clause.
- Assign  $x_i = \textit{true}$  iff row  $x_i$  selected. ▀

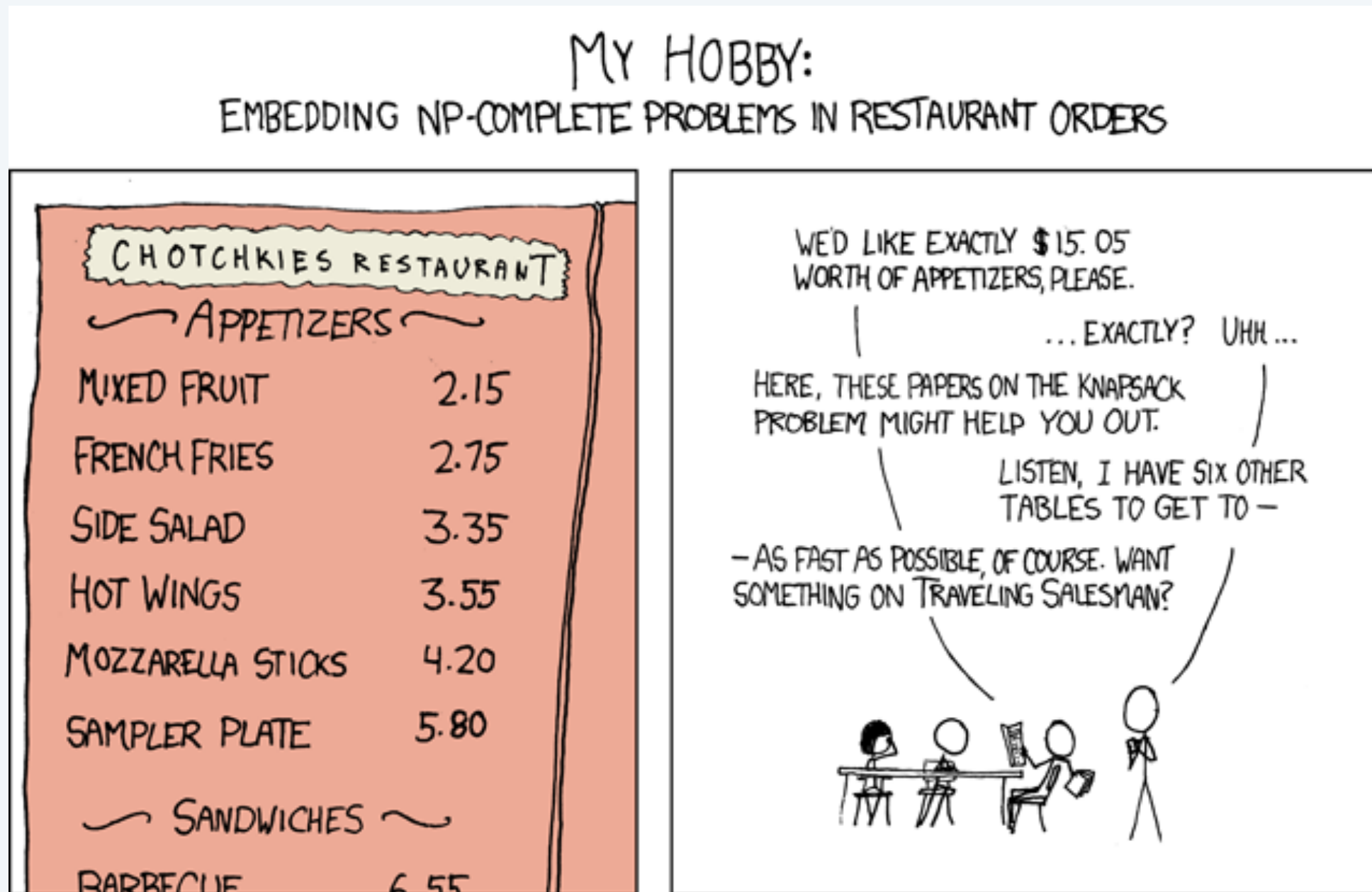
$C_1$	=	$\neg x_1$	$\vee$	$x_2$	$\vee$	$x_3$
$C_2$	=	$x_1$	$\vee$	$\neg x_2$	$\vee$	$x_3$
$C_3$	=	$\neg x_1$	$\vee$	$\neg x_2$	$\vee$	$\neg x_3$

3-SAT instance

dummies to get clause  
columns to sum to 4

	$x_1$	$x_2$	$x_3$	$C_1$	$C_2$	$C_3$	
$x_1$	1	0	0	0	1	0	100,010
$\neg x_1$	1	0	0	1	0	1	100,101
$x_2$	0	1	0	1	0	0	10,100
$\neg x_2$	0	1	0	0	1	1	10,011
$x_3$	0	0	1	1	1	0	1,110
$\neg x_3$	0	0	1	0	0	1	1,001
	0	0	0	1	0	0	100
	0	0	0	2	0	0	200
	0	0	0	0	1	0	10
	0	0	0	0	2	0	20
	0	0	0	0	0	1	1
	0	0	0	0	0	2	2
$W$	1	1	1	4	4	4	111,444

SUBSET-SUM instance



Randall Munro  
<http://xkcd.com/287>

# Partition

---

**SUBSET-SUM.** Given natural numbers  $w_1, \dots, w_n$  and an integer  $W$ , is there a subset that adds up to exactly  $W$ ?

**PARTITION.** Given natural numbers  $v_1, \dots, v_m$ , can they be partitioned into two subsets that add up to the same value  $\frac{1}{2} \sum_i v_i$ ?

**Theorem.**  $\text{SUBSET-SUM} \leq_P \text{PARTITION}$ .

**Pf.** Let  $W, w_1, \dots, w_n$  be an instance of SUBSET-SUM.

- Create instance of PARTITION with  $m = n + 2$  elements.
  - $v_1 = w_1, v_2 = w_2, \dots, v_n = w_n, v_{n+1} = 2 \sum_i w_i - W, v_{n+2} = \sum_i w_i + W$
- Lemma: there exists a subset that sums to  $W$  iff there exists a partition since elements  $v_{n+1}$  and  $v_{n+2}$  cannot be in the same partition. ■

$v_{n+1} = 2 \sum_i w_i - W$	$W$	subset A
$v_{n+2} = \sum_i w_i + W$	$\sum_i w_i - W$	subset B



# Scheduling with release times

---

**SCHEDULE.** Given a set of  $n$  jobs with processing time  $t_j$ , release time  $r_j$ , and deadline  $d_j$ , is it possible to schedule all jobs on a single machine such that job  $j$  is processed with a contiguous slot of  $t_j$  time units in the interval  $[r_j, d_j]$ ?

**Ex.**

$j$	$t_j$	$r_j$	$d_j$
1	5	0	20
2	5	8	15
3	7	0	15
4	2	10	19



# Scheduling with release times

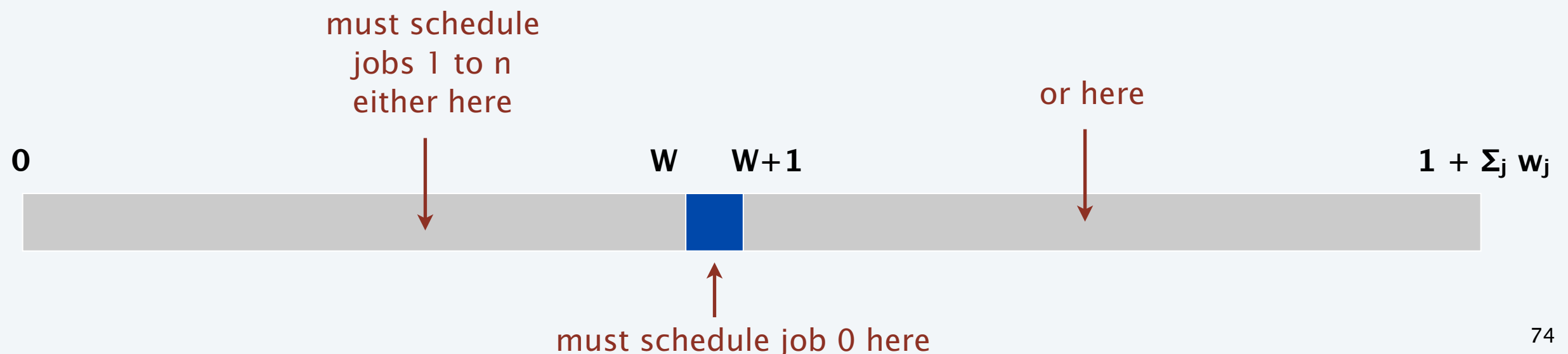
---

**Theorem.** SUBSET-SUM  $\leq_P$  SCHEDULE.

**Pf.** Given SUBSET-SUM instance  $w_1, \dots, w_n$  and target  $W$ , construct an instance of SCHEDULE that is feasible iff there exists a subset that sums to exactly  $W$ .

**Construction.**

- Create  $n$  jobs with processing time  $t_j = w_j$ , release time  $r_j = 0$ , and no deadline ( $d_j = 1 + \sum_j w_j$ ).
- Create job 0 with  $t_0 = 1$ , release time  $r_0 = W$ , and deadline  $d_0 = W + 1$ .
- Lemma: subset that sums to  $W$  iff there exists a feasible schedule. ■



# Polynomial-time reductions

---

constraint satisfaction

3-SAT

*3-SAT poly-time reduces  
to INDEPENDENT-SET*

INDEPENDENT-SET

DIR-HAM-CYCLE

GRAPH-3-COLOR

SUBSET-SUM

VERTEX-COVER

HAM-CYCLE

PLANAR-3-COLOR

SCHEDULING

SET-COVER

TSP

packing and covering

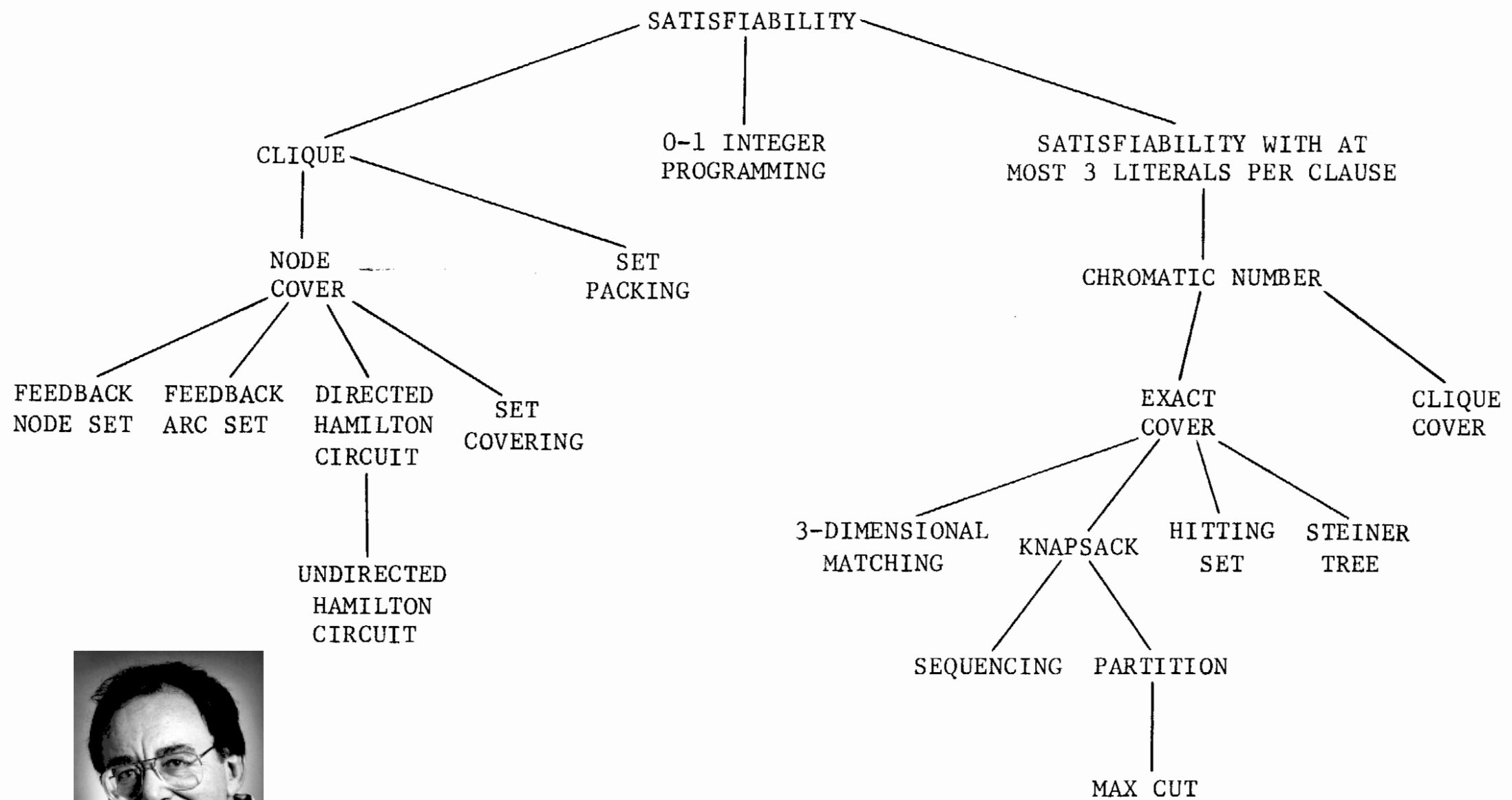
sequencing

partitioning

numerical

# Karp's 21 NP-complete problems

96



Dick Karp (1972)  
1985 Turing Award

FIGURE 1 Complete Problems

RICHARD M. KARP