CS 301

Lecture 25 – NP-complete

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April 29, 2018

Polynomial-time reducibility

A function $f : \Sigma^* \to \Sigma^*$ is a polynomial-time computable function if some poly-time TM M exists that halts with just $f(w)$ on its tape when run on w

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A is polynomial-time reducible to B (written $A \leq_{\mathrm{p}} B$) if a polynomial-time computable function f exists such that $w \in A \iff f(w) \in B$

I.e., $A \leq_{\mathrm{m}} B$ and the computable mapping is polynomial time

Another "good-news" reduction theorem

Theorem

If $A \leq_{\mathrm{p}} B$ *and* $B \in \mathrm{P}$, *then* $A \in \mathrm{P}$

Similar proof to all of the others

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Proof.

Let f be the poly-time reduction and let M decide B in poly time

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1. Compute $f(w)$
2. Run M on $f(w)$; if M accepts, then *accept*; otherwise *reject*"

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Thus, M' decides A in polynomial time so $A \in \mathrm{P}$ □

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Replacing P with NP in the proof and using NTMs rather than TMs shows that $A \leq_{\mathrm{p}} B$ and $B \in \mathrm{NP}$, then $A \in \mathrm{NP}$

CNF-SAT $\leq_p$ 3-SAT

$$\text{CNF-SAT} = \{\langle\phi\rangle \mid \phi \text{ is in CNF}\}$$
$$\text{3-SAT} = \{\langle\phi\rangle \mid \phi \text{ is in 3-CNF}\}$$

Show that $\text{CNF-SAT} \leq_p \text{3-SAT}$

To do this, we need to give a poly-time algorithm that converts ϕ in CNF to ϕ' in CNF where each clause has exactly 3 literals

$\phi = C_1 \wedge C_2 \wedge \cdots \wedge C_n$ where each C_i is a disjunction (OR) of literals

We just need a procedure to convert a clause $C = (u_1 \vee u_2 \vee \cdots \vee u_k)$ to 3-CNF

Converting a clause to 3-CNF

Four cases

1. $C = u$: Output $u_1 \vee u_1 \vee u_1$

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3. $C = u_1 \vee u_2 \vee u_3$: Output C
4. $C = u_1 \vee u_2 \vee \cdots \vee u_k$: Introduce $k-3$ new variables $z_2, z_3, \cdots z_{k-2}$ and output

$$(u_1 \vee u_2 \vee z_2) \wedge \left(\bigwedge_{i=3}^{k-2} (\overline{z_{i-1}} \vee u_i \vee z_i) \right) \wedge (\overline{z_{k-2}} \vee u_{k-1} \vee u_k)$$

For example,

$$(a \vee b \vee c \vee d \vee e) \mapsto (a \vee b \vee z_2) \wedge (\overline{z_2} \vee c \vee z_3) \wedge (\overline{z_3} \vee d \vee e)$$

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Cases 1–3 clearly preserve the property that any assignment that makes C true makes the output true and vice versa

Correctness of case 4

4. $C = u_1 \vee u_2 \vee \cdots \vee u_k$: Introduce $k-3$ new variables $z_2, z_3, \cdots z_{k-2}$ and output

$$\phi' = (u_1 \vee u_2 \vee z_2) \wedge \left(\bigwedge_{i=3}^{k-2} (\overline{z_{i-1}} \vee u_i \vee z_i) \right) \wedge (\overline{z_{k-2}} \vee u_{k-1} \vee u_k)$$

If $C = T$, then there is some true literal, say $u_i = T$, then $\phi' = T$ by setting

$$z_j = \begin{cases} T & \text{for } j < i \\ F & \text{for } j \geq i \end{cases}$$

Even if all of the other literals are false, setting z_j this way satisfies each clause in ϕ'

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- But then $(\overline{z_{k-2}} \vee u_{k-1} \vee u_k) = F$

Proof that CNF-SAT $\leq_p$ 3-SAT

Proof.

Define TM T = "On input $\langle \phi \rangle$,

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If $\langle \phi \rangle \in$ CNF-SAT, then there is some assignment of truth values to variables that makes $\phi = T$. By setting the extra variables from the algorithm appropriately, the output is satisfiable so $f(\langle \phi \rangle) \in$ 3-SAT

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If ϕ has n total literals, then the output of T has less than $3n$ clauses each of which has 3 literals so $|f(\langle\phi\rangle)| = \text{poly}(|\langle\phi\rangle|)$ thus T takes polynomial time □

NP-hard and NP-complete

Language B is NP-hard if every language $A \in \mathrm{NP}$ is poly-time reducible to B
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NP-hard and NP-complete

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Language B is NP-complete if $B \in \mathrm{NP}$ and B is NP-hard.

Equivalently, B is NP-complete if

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NP-complete problems represent the "hardest" problems in NP to solve

Any efficient solution to an NP-complete problem gives an efficient solution to every problem in NP

P, NP, and NP-complete

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If B is NP*-complete and $B \in \mathrm{P}$, then* $\mathrm{P} = \mathrm{NP}$

How would we prove this?

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If $A \in \mathrm{NP}$, then $A \leq_{\mathrm{p}} B$ and since $B \in \mathrm{P}$, $A \in \mathrm{P}$ □

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If $A \in \mathrm{NP}$, then $A \leq_{\mathrm{p}} B$ and since $B \in \mathrm{P}$, $A \in \mathrm{P}$ □

This gives us a way to try to prove that $\mathrm{P} = \mathrm{NP}$: Find an NP-complete problem and give a polynomial-time solution

Poly-time reductions between NP-complete problems

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Let $A \in \mathrm{NP}$. Because B is NP-complete, $A \leq_{\mathrm{p}} B$

Poly-time reduction is transitive and $B \leq_{\mathrm{p}} C$ so $A \leq_{\mathrm{p}} C$ thus C is NP-hard and because $C \in \mathrm{NP}$, C is NP-complete □

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Once we have one NP-complete problem, this gives us a way to find a bunch more, but we need to find one to start us off

Cook–Levin theorem

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SAT *is* NP*-complete*

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SAT *is* NP*-complete*

Sipser's proof actually shows that $\mathrm{CNF\text{-}SAT}$ is NP-complete

We showed that $\mathrm{SAT} \in \mathrm{NP}$ and a boolean formula in CNF is, of course, a boolean formula so $\langle \phi \rangle \mapsto \langle \phi \rangle$ is polynomial-time reduction from $\mathrm{CNF\text{-}SAT}$ to SAT

3-SAT is NP-complete

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To prove this, we need to show two things: 3-SAT $\in$ NP and there is some NP-complete language A that poly-time reduces to 3-SAT

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To prove this, we need to show two things: $\text{3-SAT} \in \text{NP}$ and there is some NP-complete language A that poly-time reduces to 3-SAT

Proof.

We already showed that $\text{CNF-SAT} \leq_{\text{p}} \text{3-SAT}$ so all that remains is to show that $\text{3-SAT} \in \text{NP}$

But this is true for the same reason $\text{SAT} \in \text{NP}$: We can verify an assignment of truth values to variables satisfies a formula in poly time □

General technique

If we want to show that some language L is NP-complete, then we need to perform 3 steps

1. Show that $L \in \mathrm{NP}$
2. Select some NP-complete language B
3. Show that $B \leq_{\mathrm{p}} L$

Step 1 is frequently easy: If the language is of the form $\{w \mid w \text{ has some property or structure}\}$, then the certificate for the verifier is whatever makes the property true or the structure itself

Steps 2 and 3 can be quite challenging and can require a great deal of cleverness; $3\text{-}\mathrm{SAT}$ is usually a good first choice for the NP-complete language

VertexCover is NP-complete

Recall VertexCover = $\{\langle G, k\rangle \mid G$ has a vertex cover of size $k\} \in \mathrm{NP}$ because the vertex cover itself is the certificate

To show that VertexCover is NP-complete, we want to give a poly-time reduction from 3-SAT

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To do this, we want to take a formula ϕ that has m clauses $C_1, C_2, \ldots, C_m$ and n variables $x_1, x_2, \ldots, x_n$ and construct an undirected graph $G = (V, E)$ and a k s.t. G has a vertex cover of size k iff ϕ is satisfiable

That is, we need to produce a mapping $\langle\phi\rangle \mapsto \langle G, k\rangle$ such that $\langle\phi\rangle \in$ 3-SAT $\iff \langle G, k\rangle \in$ VertexCover and we have to be able to compute the mapping in some polynomial of m and n

Gadgets

For each variable and each clause, we want to construct some portion of a graph

Running example: $\phi = (\underbrace{x_1 \vee \overline{x_2} \vee x_3}_{C_1}) \wedge (\underbrace{\overline{x_1} \vee \overline{x_2} \vee \overline{x_3}}_{C_2})$

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1. **Assignment.** For each variable x_i create vertices x_i and $\overline{x_i}$ and edge $(x_i, \overline{x_i})$

$$V_A = \bigcup_{i=1}^{n} \{x_i, \overline{x_i}\}$$

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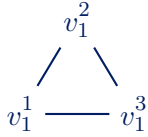

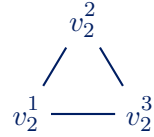

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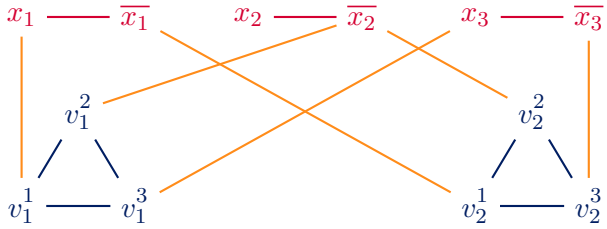

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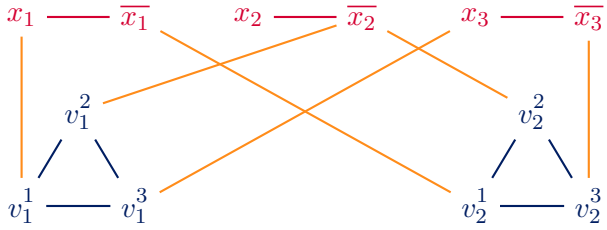

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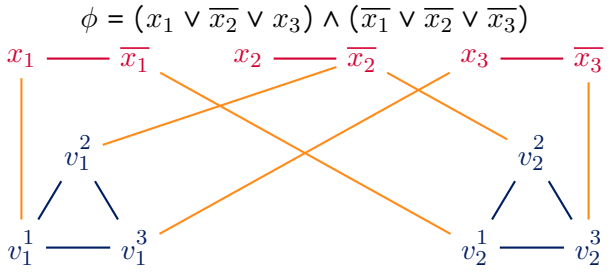
Output: $G = (V, E)$, k where

$V = V_A \cup V_S$

$E = E_A \cup E_S \cup E_C$

$k = n + 2m$

If G has a VC of size $n + 2m \ldots$



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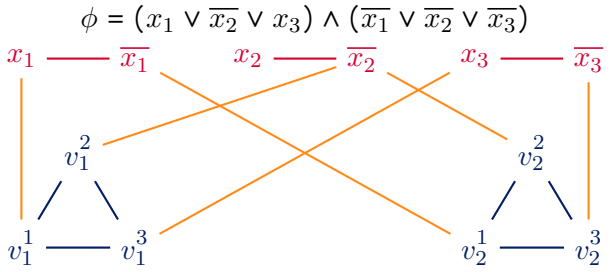

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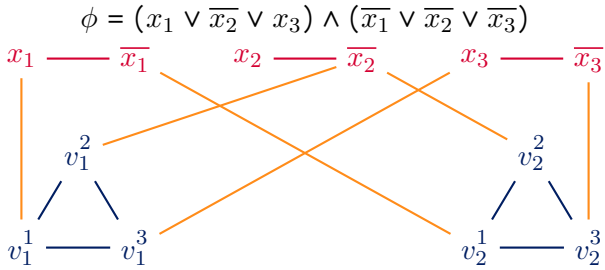

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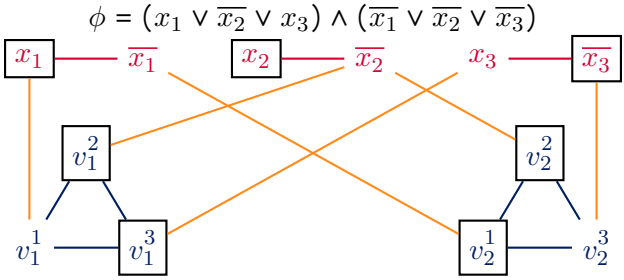

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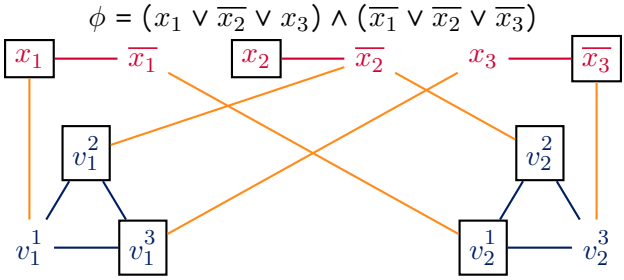
For example, the boxed vertices form a vertex cover of size $n + 2m = 7$

If G has a VC of size $n + 2m$, then ϕ is satisfiable



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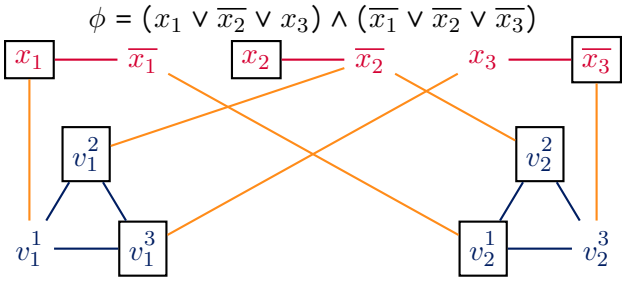

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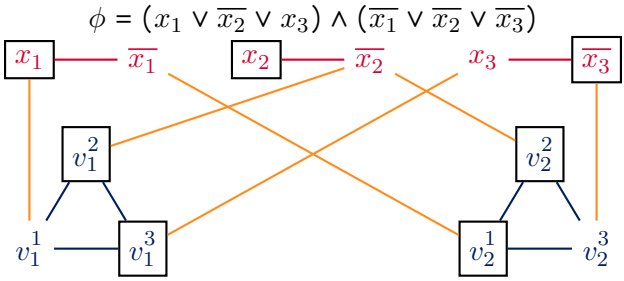

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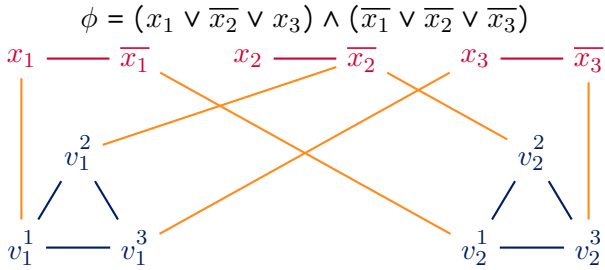

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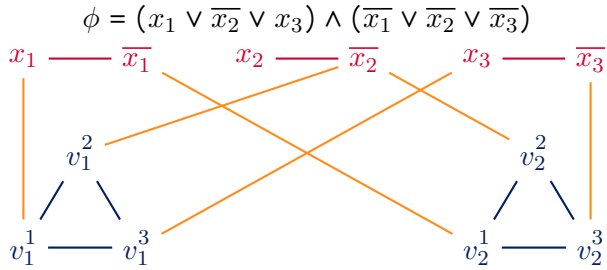

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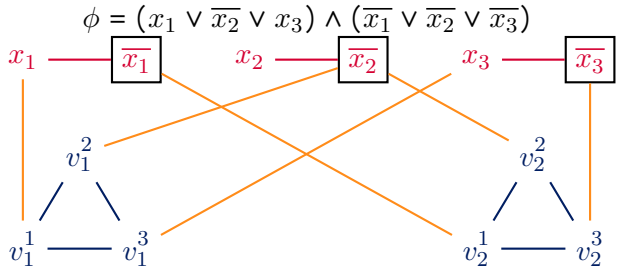

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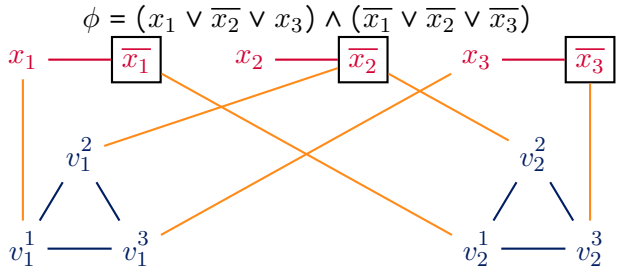

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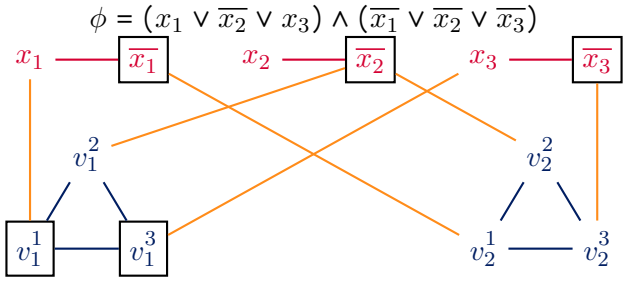

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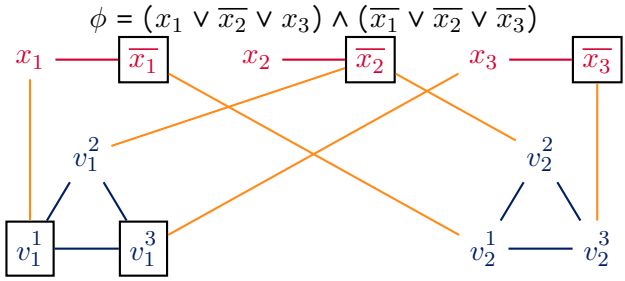

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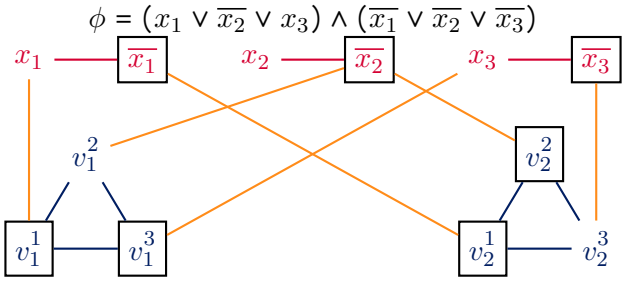

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Lastly, G takes polynomial time to create since it has a polynomial number of vertices and edges

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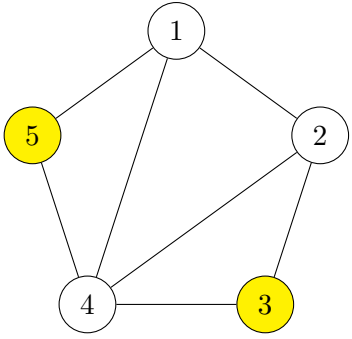
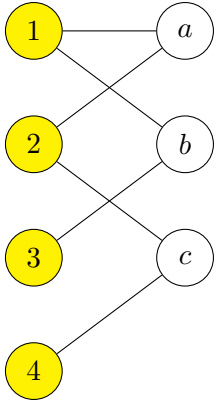
Steps 2–5 show 3-SAT $\leq_{\mathrm{p}}$ VertexCover and thus VertexCover is NP-complete

Independent set

If $G = (V, E)$ is an undirected graph, an independent set is a set $I \subseteq V$ such that no two vertices in I are adjacent
I.e., $\forall u, v \in I \ (u, v) \notin E$

E.g., the yellow vertices form an independent set





IndSet

IndSet = $\{\langle G, k\rangle \mid G$ is an undirected graph which has an independent set of size $k\}$

How would we show that IndSet is NP-complete?

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How would we show that IndSet is NP-complete?

We need to show

1. IndSet $\in$ NP and
2. There is some A which is NP-complete and $A \leq_{\mathrm{p}}$ IndSet

IndSet ∈ NP

What is a certificate for IndSet?

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IndSet ∈ NP

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We can build a verifier for IndSet:
V = "On input $\langle G, k, I \rangle$ where $G = (V, E)$,

1. If $I \nsubseteq V$ or $|I| \neq k$, then *reject*
2. For each $(u, v) \in E$,
3. If $u \in I$ and $v \in I$, then *reject*
4. Otherwise *accept*"

Each step clearly takes polynomial time and the body of the loop happens once per edge so V is a polynomial time verifier

VertexCover $\leq_p$ IndSet

We can reduce from VertexCover to IndSet by giving a polynomial time map $\langle G, k \rangle \mapsto \langle G', k' \rangle$ such that
$\langle G, k \rangle \in$ VertexCover $\iff \langle G', k' \rangle \in$ IndSet

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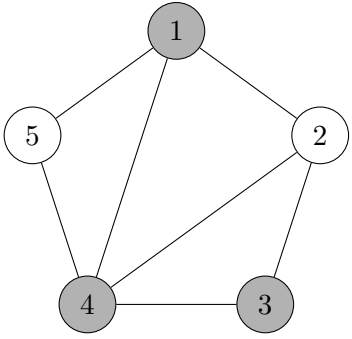

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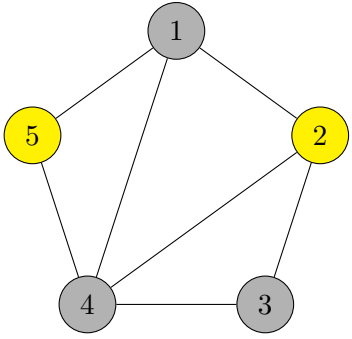

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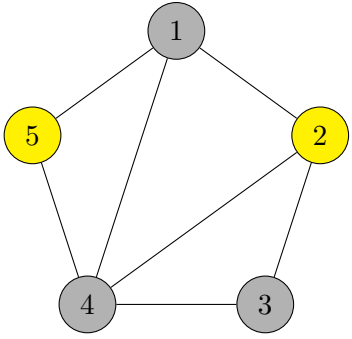

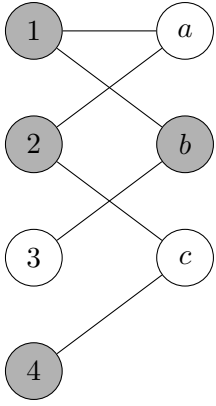

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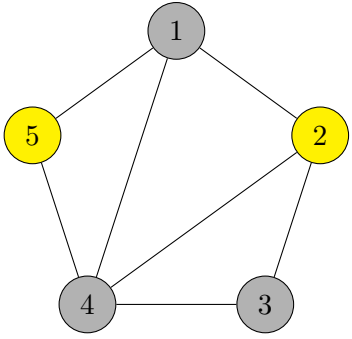

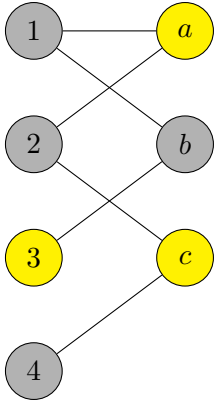

Relationship between vertex covers and independent sets

It *looks* like if $G = (V, E)$ has a vertex cover C, then $I = V \setminus C$ is an independent set, and vice versa

Can we prove that?

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How does this help us?

It means that G has n vertices, then G has a vertex cover of size k iff G has an independent set of size $n - k$

VertexCover $\leq_p$ IndSet

Proof.

Our polynomial time mapping is $\langle G, k \rangle \mapsto \langle G, n - k \rangle$ where $G = (V, E)$ and $|V| = n$

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Since G has a vertex cover of size k iff G has an independent set of size $n - k$,

$$\langle G, k \rangle \in \text{VertexCover} \iff \langle G, n - k \rangle \in \text{IndSet}$$

Since IndSet $\in$ NP, VertexCover $\leq_p$ IndSet, and VertexCover is NP-complete, IndSet is NP-complete $\square$

Clique is NP-complete

We already proved that Clique $\in$ NP so all that remains is to give a polynomial time mapping from some NP-complete problem

Let's use IndSet

CLIQUE is NP-complete

We already proved that CLIQUE $\in$ NP so all that remains is to give a polynomial time mapping from some NP-complete problem

Let's use INDSET

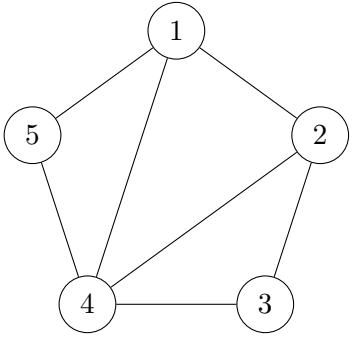
We want a mapping $\langle G, k \rangle \mapsto \langle G', k' \rangle$ such that G has an independent set of size k iff G' has a clique of size k'

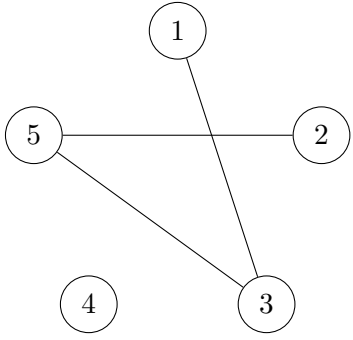
Recall

- **Independent set.** I is an independent set if there is no edge between any two vertices in I
- **Clique.** C is a clique if there is an edge between every two vertices in C

Complement of a graph

The complement of a graph $G = (V, E)$ is a graph $G' = (V, E')$ where $(u, v) \in E$ iff $(u, v) \notin E'$ (assuming $u \neq v$)





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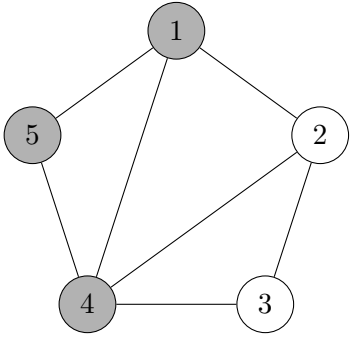
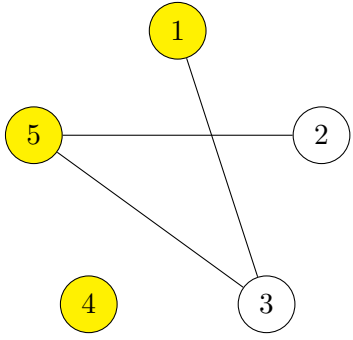
Grey vertices form a clique, yellow form an independent set





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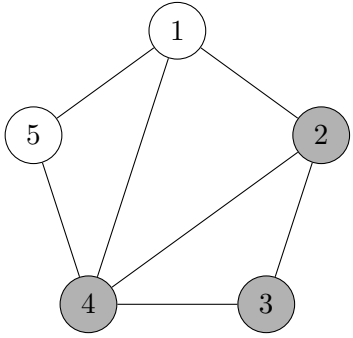
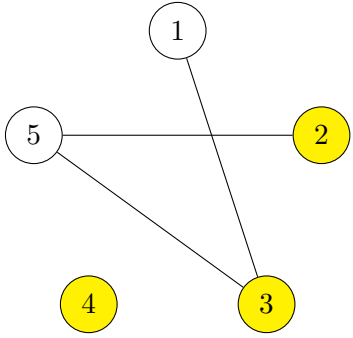
The complement of a graph $G = (V, E)$ is a graph $G' = (V, E')$ where $(u, v) \in E$ iff $(u, v) \notin E'$ (assuming $u \neq v$)

Grey vertices form a clique, yellow form an independent set





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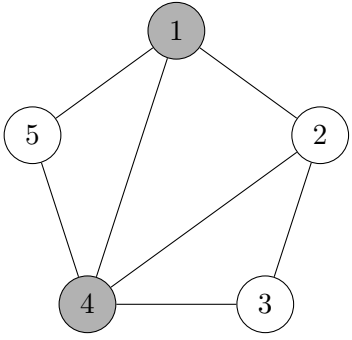

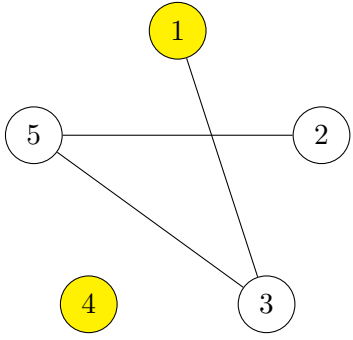

Relationship between a clique and an independent set

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- And vice versa

IndSet $\leq_p$ Clique

The polynomial time mapping is $\langle G, k \rangle \mapsto \langle G', k \rangle$ where G' is the complement of G

Since Clique $\in$ NP and IndSet $\leq_p$ Clique, Clique is NP-complete

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