

Introduction to solving intractable problems

Serge Gaspers

UNSW

Outline

- 1 Algorithms for NP-hard problems
- 2 Exponential Time Algorithms
- 3 Parameterized Complexity
 - FPT Algorithm for Vertex Cover
 - Algorithms for Vertex Cover
- 4 Further Reading

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Central question

P vs. NP

NP-hard problems

- no known polynomial time algorithm for any NP-hard problem
- belief: $P \neq NP$
- What to do when facing an NP-hard problem?

Example problem

Monitoring a power grid

Tammy is responsible for fault detection on the power grid of an energy company. She has access to k monitoring devices. Each one can be placed on a node of the electrical grid and can monitor the power lines that are connected to this node. Tammy's objective is to place the monitoring devices in such a way that each power line is monitored by at least one monitoring device.

Let us first give an abstraction of this problem and formulate it as a decision problem for graphs.

Example problem: VERTEX COVER

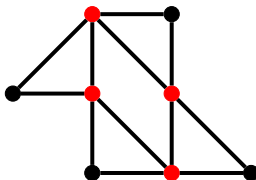
A **vertex cover** in a graph $G = (V, E)$ is a subset of vertices $S \subseteq V$ such that every edge of G has an endpoint in S .

VERTEX COVER

Input: Graph G , integer k

Question: Does G have a vertex cover of size k ?

Note: VERTEX COVER is **NP**-complete.



Coping with NP-hardness

- Approximation algorithms
 - There is a polynomial-time algorithm, which, given a graph G , finds a vertex cover of G of size at most $2 \cdot \text{OPT}$, where OPT is the size of a smallest vertex cover of G .
- Exact exponential time algorithms
 - There is an algorithm solving VERTEX COVER in time $O(1.1970^n)$, where $n = |V|$ (Xiao and Nagamochi, 2017).
- Fixed parameter algorithms
 - There is an algorithm solving VERTEX COVER in time $O(1.2738^k + kn)$ (Chen, Kanj, and Xia, 2010).
- Heuristics
 - The COVER heuristic (COVer Edges Randomly) finds a smaller vertex cover than state-of-the-art heuristics on a suite of hard benchmark instances (Richter, Helmert, and Gretton, 2007).
- Restricting the inputs
 - VERTEX COVER can be solved in polynomial time on bipartite graphs, trees, interval graphs, etc. (Golumbic, 2004).
- Quantum algorithms?
 - Not believed to solve NP-hard problems in polynomial time (Aaronson, 2005). Quadratic speedup possible in some cases.

Aims of this course

Design and analyze algorithms for NP-hard problems.

We focus on algorithms that solve NP-hard problems **exactly** and analyze their **worst case running time**.

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Running times

Worst case running time of an algorithm.

- An algorithm is **polynomial** if $\exists c \in \mathbb{N}$ such that the algorithm solves every instance in time $O(n^c)$, where n is the size of the instance.

Also: $n^{O(1)}$ or **poly**(n).

- **quasi-polynomial**: $2^{O(\log^c n)}$, $c \in O(1)$
- **sub-exponential**: $2^{o(n)}$
- **exponential**: $2^{\text{poly}(n)}$
- **double-exponential**: $2^{2^{\text{poly}(n)}}$

O^* -notation ignores polynomial factors in the input size:

$$O^*(f(n)) \equiv O(f(n) \cdot \text{poly}(n))$$

$$O^*(f(k)) \equiv O(f(k) \cdot \text{poly}(n))$$

Brute-force algorithms for NP-hard problems

Theorem 1

Every problem in NP can be solved in exponential time.

Brute-force algorithms for NP-hard problems

Theorem 1

Every problem in NP can be solved in exponential time.

For a proof, see the lecture on NP-completeness.

Three main categories for NP-complete problems

- Subset problems
- Permutation problems
- Partition problems

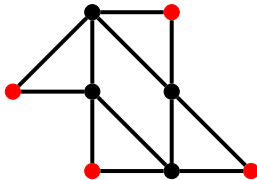
Subset Problem: INDEPENDENT SET

An **independent set** in a graph $G = (V, E)$ is a subset of vertices $S \subseteq V$ such that the vertices in S are pairwise non-adjacent in G .

INDEPENDENT SET

Input: Graph G , integer k

Question: Does G have an independent set of size k ?



Brute-force:

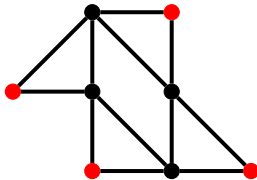
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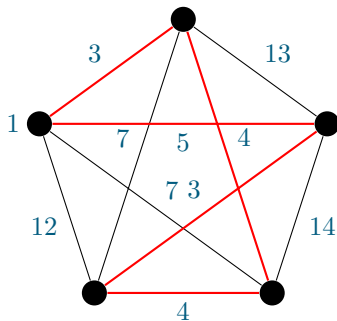
Brute-force: $O^*(2^n)$, where $n = |V(G)|$

Permutation Problem: TRAVELING SALESPERSON

TRAVELING SALESPERSON (TSP)

Input: a set of n cities, the distance $d(i, j) \in \mathbb{N}$ between every two cities i and j , integer k

Question: Is there a permutation of the cities (a **tour**) such that the total distance when traveling from city to city in the specified order, and returning back to the origin, is at most k ?



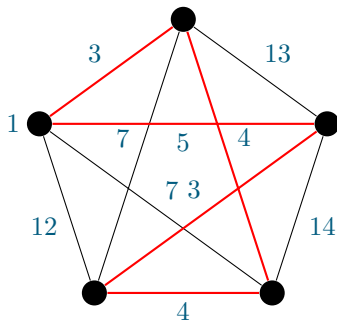
Brute-force:

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Brute-force: $O^*(n!) \subseteq 2^{O(n \log n)}$

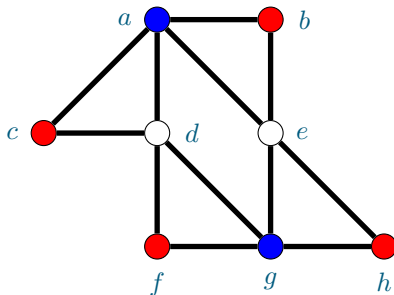
Partition Problem: COLORING

A k -coloring of a graph $G = (V, E)$ is a function $f : V \rightarrow \{1, 2, \dots, k\}$ assigning colors to V such that no two adjacent vertices receive the same color.

COLORING

Input: Graph G , integer k

Question: Does G have a k -coloring?



Brute-force:

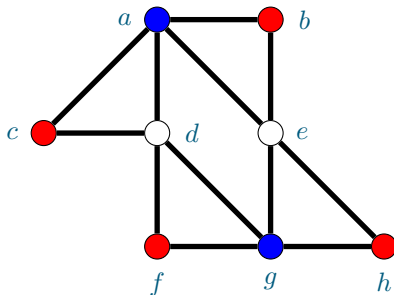
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COLORING

Input: Graph G , integer k

Question: Does G have a k -coloring?



Brute-force: $O^*(k^n)$, where $n = |V(G)|$

Exponential Time Algorithms

- natural question in Algorithms:
design faster (worst-case analysis) algorithms for problems
- might lead to practical algorithms
 - for small instances
 - you don't want to design software where your client/boss can find with better solutions *by hand* than your software
 - subroutines for
 - (sub)exponential time approximation algorithms
 - randomized algorithms with expected polynomial run time

Solve an NP-hard problem

- exhaustive search
 - trivial method
 - try all candidate solutions (certificates) for a ground set on n elements
 - running times for problems in NP
 - SUBSET PROBLEMS: $O^*(2^n)$
 - PERMUTATION PROBLEMS: $O^*(n!)$
 - PARTITION PROBLEMS: $O^*(c^{n \log n})$
- faster exact algorithms
 - for some problems, it is possible to obtain provably faster algorithms
 - running times $O(1.0836^n)$, $O(1.4689^n)$, $O(1.9977^n)$

Exponential Time Algorithms in Practice

- How large are the instances one can solve in practice?

| Available time nb. of operations | 1 s 2^{38} | 1 min $\sim 2^{44}$ | 1 hour $\sim 2^{50}$ | 3 days $\sim 2^{56}$ | 6 months $\sim 2^{62}$ |
|-------------------------------------|-----------------|------------------------|-------------------------|-------------------------|---------------------------|
| n^5 | 194 | 446 | 1,024 | 2,352 | 5,404 |
| n^{10} | 14 | 21 | 32 | 49 | 74 |
| 1.05^n | 540 | 625 | 711 | 796 | 881 |
| 1.1^n | 276 | 320 | 364 | 407 | 451 |
| 1.5^n | 65 | 75 | 85 | 96 | 106 |
| 2^n | 38 | 44 | 50 | 56 | 62 |
| 5^n | 16 | 19 | 22 | 24 | 27 |
| $n!$ | 14 | 16 | 17 | 19 | 20 |

Note: Intel Core i7-8086K executes $\sim 2^{38}$ instructions per second at 5 GHz.

“For every polynomial-time algorithm you have, there is an exponential algorithm that I would rather run.”

– Alan Perlis (1922-1990, programming languages, 1st recipient of Turing Award)

Hardware vs. Algorithms

- Suppose a 2^n algorithm enables us to solve instances up to size x
- Faster processors
 - processor speed doubles after 18–24 months (Moore's law)
 - can solve instances up to size $x + 1$
- Faster algorithm
 - design an $O^*(2^{n/2}) \subseteq O(1.4143^n)$ time algorithm
 - can solve instances up to size $2 \cdot x$

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A story

A computer scientist meets a biologist . . .

Eliminating conflicts from experiments

$n = 1000$ experiments,
 $k = 20$ experiments failed

| Theoretical | Running Time | |
|---------------|------------------------|------------------------------|
| | Number of Instructions | Real |
| 2^n | $1.07 \cdot 10^{301}$ | $4.941 \cdot 10^{282}$ years |
| n^k | 10^{60} | $4.611 \cdot 10^{41}$ years |
| $2^k \cdot n$ | $1.05 \cdot 10^9$ | 0.01526 seconds |

Notes

- We assume that 2^{36} instructions are carried out per second.
- The Big Bang happened roughly $13.5 \cdot 10^9$ years ago.

Goal of Parameterized Complexity

Confine the combinatorial explosion to a parameter k .



For which problem–parameter combinations can we find algorithms with running times of the form

$$f(k) \cdot n^{O(1)},$$

where the f is a computable function independent of the input size n ?

Examples of Parameters

A Parameterized Problem

Input: an instance of the problem

Parameter: a parameter k

Question: a YES/NO question about the instance and the parameter

- A parameter can be
 - input size (trivial parameterization)
 - solution size
 - related to the structure of the input (maximum degree, treewidth, branchwidth, genus, ...)
 - etc.

Main Complexity Classes

P: class of problems that can be solved in time $n^{O(1)}$

FPT: class of problems that can be solved in time $f(k) \cdot n^{O(1)}$

W[·]: parameterized intractability classes

XP: class of problems that can be solved in time $f(k) \cdot n^{g(k)}$

$$P \subseteq FPT \subseteq W[1] \subseteq W[2] \cdots \subseteq W[P] \subseteq XP$$

Known: If $FPT = W[1]$, then the Exponential Time Hypothesis fails, i.e. 3-SAT can be solved in time $2^{o(n)}$.

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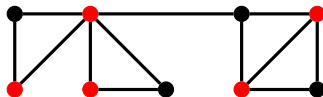
Vertex Cover

VERTEX COVER (VC)

Input: A graph $G = (V, E)$ on n vertices, an integer k

Parameter: k

Question: Is there a set of vertices $C \subseteq V$ of size at most k such that every edge has at least one endpoint in C ?



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Brute Force Algorithms

- $2^n \cdot n^{O(1)}$ not FPT
- $n^k \cdot n^{O(1)}$ not FPT

An FPT Algorithm

Algorithm $\text{vc1}(G, k)$;

```
1 if  $E = \emptyset$  then                // all edges are covered
2   | return Yes
3 else if  $k \leq 0$  then              // we cannot select any vertex
4   | return No
5 else
6   | Select an edge  $uv \in E$ ;
7   | return  $\text{vc1}(G - u, k - 1) \vee \text{vc1}(G - v, k - 1)$ 
```

Running Time Analysis

- Let us look at an arbitrary execution of the algorithm.
- Recursive calls form a **search tree** T
 - with depth $\leq k$
 - where each node has ≤ 2 children
- $\Rightarrow T$ has $\leq 2^k$ leaves and $\leq 2^k - 1$ internal nodes
- at each node the algorithm spends time $n^{O(1)}$
- The running time is $O^*(2^k)$

A faster FPT Algorithm

A faster FPT Algorithm

Algorithm $\text{vc2}(G, k)$;

```
1 if  $E = \emptyset$  then                                // all edges are covered
2   | return Yes
3 else if  $k \leq 0$  then                               // we used too many vertices
4   | return No
5 else if  $\Delta(G) \leq 2$  then                         //  $G$  has maximum degree  $\leq 2$ 
6   | Solve the problem in polynomial time;
7 else
8   | Select a vertex  $v$  of maximum degree;
9   | return  $\text{vc2}(G - v, k - 1) \vee \text{vc2}(G - N[v], k - d(v))$ 
```

Running time analysis of vc2

- Number of leaves of the search tree:

$$T(k) \leq T(k-1) + T(k-3)$$

$$x^k \leq x^{k-1} + x^{k-3}$$

$$x^3 - x^2 - 1 \leq 0$$

- The equation $x^3 - x^2 - 1 = 0$ has a unique positive real solution:
 $x \approx 1.4655 \dots$
- Running time: $1.4656^k \cdot n^{O(1)}$

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- Exponential-time algorithms
 - Chapter 1, *Introduction*, in (Fomin and Kratsch, 2010).
 - Survey on exponential-time algorithms (Woeginger, 2001).
 - Chapter 1, *Introduction*, in (Gaspers, 2010).
- Parameterized Complexity
 - Chapter 1, *Introduction*, in (Cygan et al., 2015)
 - Chapter 2, *The Basic Definitions*, in (Downey and Fellows, 2013)
 - Chapter I, *Foundations*, in (Niedermeier, 2006)
 - *Preface* in (Flum and Grohe, 2006)

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NP-completeness

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Polynomial time

Polynomial-time algorithm

Polynomial-time algorithm:

There exists a constant $c \in \mathbb{N}$ such that the algorithm has (worst-case) running-time $O(n^c)$, where n is the size of the input.

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Example

Polynomial: n ; $n^2 \log_2 n$; n^3 ; n^{20}

Super-polynomial: $n^{\log_2 n}$; $2^{\sqrt{n}}$; 1.001^n ; 2^n ; $n!$

Tractable problems

Central Question

Which computational problems have polynomial-time algorithms?

Million-dollar question

Intriguing class of problems: NP-complete problems.

NP-complete problems

It is unknown whether NP-complete problems have polynomial-time algorithms.

- A polynomial-time algorithm for one NP-complete problem would imply polynomial-time algorithms for all problems in NP.

Gerhard Woeginger's P vs NP page:

<http://www.win.tue.nl/~gwoegi/P-versus-NP.htm>

Polynomial vs. NP-complete

Polynomial

- SHORTEST PATH: Given a graph G , two vertices a and b of G , and an integer k , does G have a simple a - b -path of length at most k ?
- EULER TOUR: Given a graph G , does G have a cycle that traverses each edge of G exactly once?
- 2-CNF SAT: Given a propositional formula F in 2-CNF, is F satisfiable?

A k -CNF formula is a conjunction (AND) of clauses, and each clause is a disjunction (OR) of at most k literals, which are negated or unnegated Boolean variables.

NP-complete

- LONGEST PATH: Given a graph G and an integer k , does G have a simple path of length at least k ?
- HAMILTONIAN CYCLE: Given a graph G , does G have a simple cycle that visits each vertex of G ?
- 3-CNF SAT: Given a propositional formula F in 3-CNF, is F satisfiable?

Example:

$$(x \vee \neg y \vee z) \wedge (\neg x \vee z) \wedge (\neg y \vee \neg z).$$

What's next?

- Formally define P , NP , and NP -complete (NPC)
- (New) skill: show that a problem is NP -complete

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Decision problems and Encodings

<Name of Decision Problem>

Input: <What constitutes an instance>

Question: <Yes/No question>

Decision problems and Encodings

<Name of Decision Problem>

Input: <What constitutes an instance>

Question: <Yes/No question>

We want to know which decision problems can be solved in polynomial time – polynomial in the **size of the input** n .

- Assume a “reasonable” encoding of the input
- Many encodings are polynomial-time equivalent; i.e., one encoding can be computed from another in polynomial time.
- Important exception: unary versus binary encoding of integers.
 - An integer x takes $\lceil \log_2 x \rceil$ bits in binary and $x = 2^{\log_2 x}$ bits in unary.

Formal-language framework

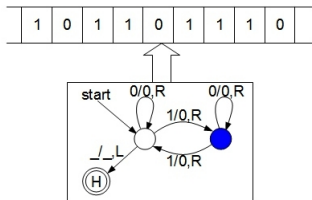
We can view decision problems as languages.

- Alphabet Σ : finite set of symbols. W.l.o.g., $\Sigma = \{0, 1\}$
- Language L over Σ : set of strings made with symbols from Σ : $L \subseteq \Sigma^*$
- Fix an encoding of instances of a decision problem Π into Σ
- Define the language $L_\Pi \subseteq \Sigma^*$ such that

$$x \in L_\Pi \Leftrightarrow x \text{ is a Yes-instance for } \Pi$$

Non-deterministic Turing Machine (NTM)

- **input word** $x \in \Sigma^*$ placed on an **infinite tape** (memory)
- read-write head initially placed on the first symbol of x
- computation step: if the machine is in state s and reads a , it can move into state s' , writing b , and moving the head into direction $D \in \{L, R\}$ if $((s, a), (s', b, D)) \in \delta$.



- Q : finite, non-empty set of states
- Γ : finite, non-empty set of tape symbols
- $_ \in \Gamma$: blank symbol (the only symbol allowed to occur on the tape infinitely often)
- $\Sigma \subseteq \Gamma \setminus \{b\}$: set of input symbols
- $q_0 \in Q$: start state
- $A \subseteq Q$: set of accepting (final) states
- $\delta \subseteq (Q \setminus A \times \Gamma) \times (Q \times \Gamma \times \{L, R\})$: transition relation, where L stands for a move to the left and R for a move to the right.

Definition 1

A NTM **accepts** a word $x \in \Sigma^*$ if there exists a sequence of computation steps starting in the start state and ending in an accept state.

Definition 2

The language **accepted** by an NTM is the set of words it accepts.

Acceptance in polynomial time

Definition 3

A language L is **accepted in polynomial time** by an NTM M if

- L is accepted by M , and
- there is a constant k such that for any word $x \in L$, the NTM M accepts x in $O(|x|^k)$ computation steps.

Deterministic Turing Machine

Definition 4

A **Deterministic Turing Machine (DTM)** is a Non-deterministic Turing Machine where the transition relation contains at most one tuple $((s, a), (\cdot, \cdot, \cdot))$ for each $s \in Q \setminus A$ and $a \in \Gamma$.

The transition relation δ can be viewed as a function

$$\delta : Q \setminus A \times \Gamma \rightarrow Q \times \Gamma \times \{L, R\}.$$

\Rightarrow For a given input word $x \in \Sigma^*$, there is exactly one sequence of computation steps starting in the start state.

Many computational models are polynomial-time equivalent to DTMs:

- Random Access Machine (RAM, used for algorithms in the textbook)
- variants of Turing machines (multiple tapes, infinite only in one direction, ...)
- ...

Definition 5 (P)

$P = \{L \subseteq \Sigma^* : \text{there is a DTM accepting } L \text{ in polynomial time}\}$

Definition 6 (NP)

$NP = \{L \subseteq \Sigma^* : \text{there is a NTM accepting } L \text{ in polynomial time}\}$

Definition 7 (coNP)

$coNP = \{L \subseteq \Sigma^* : \Sigma^* \setminus L \in NP\}$

Theorem 8

If $L \in P$, then there is a polynomial-time DTM that halts in an accepting state on every word in L and it halts in a non-accepting state on every word not in L .

Theorem 8

If $L \in \mathbf{P}$, then there is a polynomial-time DTM that halts in an accepting state on every word in L and it halts in a non-accepting state on every word not in L .

Proof sketch.

Suppose $L \in \mathbf{P}$. By the definition of \mathbf{P} , there is a DTM M that accepts L in polynomial time.

Idea: design a DTM M' that simulates M for $c \cdot n^k$ steps, where $c \cdot n^k$ is the running time of M and transitions to a non-accepting state if M does not halt in an accepting state.

(Note that this proof is nonconstructive: we might not know the running time of M .) □

Non-deterministic choices

A NTM for an NP-language L makes a polynomial number of non-deterministic choices on input $x \in L$.

We can encode these non-deterministic choices into a certificate c , which is a polynomial-length word.

Now, there exists a DTM, which, given x and c , verifies that $x \in L$ in polynomial time.

Thus, $L \in \text{NP}$ iff there is a DTM V and for each $x \in L$ there exists a polynomial-length certificate c such that $V(x, c) = 1$, but $V(y, \cdot) = 0$ for each $y \notin L$.

CNF-SAT is in NP

- A **CNF formula** is a propositional formula in conjunctive normal form: a conjunction (AND) of clauses; each clause is a disjunction (OR) of literals; each literal is a negated or unnegated Boolean variable.
- An assignment $\alpha : \text{var}(F) \rightarrow \{0, 1\}$ satisfies a clause C if it sets a literal of C to true, and it satisfies F if it satisfies all clauses in F .

CNF-SAT

Input: CNF formula F

Question: Does F have a satisfying assignment?

Example: $(x \vee \neg y \vee z) \wedge (\neg x \vee z) \wedge (\neg y \vee \neg z)$.

Lemma 9

CNF-SAT \in **NP**.

CNF-SAT is in NP

- A **CNF formula** is a propositional formula in conjunctive normal form: a conjunction (AND) of clauses; each clause is a disjunction (OR) of literals; each literal is a negated or unnegated Boolean variable.
- An assignment $\alpha : \text{var}(F) \rightarrow \{0, 1\}$ satisfies a clause C if it sets a literal of C to true, and it satisfies F if it satisfies all clauses in F .

CNF-SAT

Input: CNF formula F

Question: Does F have a satisfying assignment?

Example: $(x \vee \neg y \vee z) \wedge (\neg x \vee z) \wedge (\neg y \vee \neg z)$.

Lemma 9

CNF-SAT \in **NP**.

Proof.

Certificate: assignment α to the variables.

Given a certificate, it can be checked in polynomial time whether all clauses are satisfied. □

Brute-force algorithms for problems in NP

Theorem 10

Every problem in NP can be solved in exponential time.

Brute-force algorithms for problems in NP

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Every problem in NP can be solved in exponential time.

Proof.

Let Π be an arbitrary problem in NP. [Use certificate-based definition of NP]

We know that \exists a polynomial p and a polynomial-time verification algorithm V such that:

- for every $x \in \Pi$ (i.e., every YES-instance for Π) \exists string $c \in \{0, 1\}^*$, $|c| \leq p(|x|)$, such that $V(x, c) = 1$, and
- for every $x \notin \Pi$ (i.e., every NO-instance for Π) and every string $c \in \{0, 1\}^*$, $V(x, c) = 0$.

Brute-force algorithms for problems in NP

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Every problem in **NP** can be solved in exponential time.

Proof.

Let Π be an arbitrary problem in **NP**. [Use certificate-based definition of **NP**]
We know that \exists a polynomial p and a polynomial-time verification algorithm V such that:

- for every $x \in \Pi$ (i.e., every **YES**-instance for Π) \exists string $c \in \{0, 1\}^*$, $|c| \leq p(|x|)$, such that $V(x, c) = 1$, and
- for every $x \notin \Pi$ (i.e., every **NO**-instance for Π) and every string $c \in \{0, 1\}^*$, $V(x, c) = 0$.

Now, we can prove there exists an exponential-time algorithm for Π with input x :

- For each string $c \in \{0, 1\}^*$ with $|c| \leq p(|x|)$, evaluate $V(x, c)$ and return **YES** if $V(x, c) = 1$.
- Return **NO**.

Running time: $2^{p(|x|)} \cdot n^{O(1)} \subseteq 2^{O(2 \cdot p(|x|))} = 2^{O(p(|x|))}$, but non-constructive. \square

Outline

- 1 Overview
- 2 Turing Machines, P, and NP
- 3 Reductions and NP-completeness**
- 4 NP-complete problems
- 5 Further Reading

Polynomial-time reduction

Definition 11

A language L_1 is **polynomial-time reducible** to a language L_2 , written $L_1 \leq_P L_2$, if there exists a polynomial-time computable function $f : \Sigma^* \rightarrow \Sigma^*$ such that for all $x \in \Sigma^*$,

$$x \in L_1 \Leftrightarrow f(x) \in L_2.$$

A polynomial time algorithm computing f is a **reduction algorithm**.

New polynomial-time algorithms via reductions

Lemma 12

If $L_1, L_2 \in \Sigma^$ are languages such that $L_1 \leq_P L_2$, then $L_2 \in \mathbf{P}$ implies $L_1 \in \mathbf{P}$.*

Definition 13 (NP-hard)

A language $L \subseteq \Sigma^*$ is **NP-hard** if

$$L' \leq_P L \text{ for every } L' \in \text{NP}.$$

Definition 14 (NP-complete)

A language $L \subseteq \Sigma^*$ is **NP-complete** (in **NPC**) if

- 1 $L \in \text{NP}$, and
- 2 L is **NP-hard**.

A first NP-complete problem

Theorem 15

CNF-SAT is NP-complete.

Proved by encoding NTMs into SAT (Cook, 1971; Levin, 1973) and then CNF-SAT (Karp, 1972).

Proving NP-completeness

Lemma 16

*If L is a language such that $L' \leq_P L$ for some $L' \in \text{NPC}$, then L is NP-hard.
If, in addition, $L \in \text{NP}$, then $L \in \text{NPC}$.*

Proving NP-completeness

Lemma 16

If L is a language such that $L' \leq_P L$ for some $L' \in \text{NPC}$, then L is NP-hard.
If, in addition, $L \in \text{NP}$, then $L \in \text{NPC}$.

Proof.

For all $L'' \in \text{NP}$, we have $L'' \leq_P L' \leq_P L$.

By transitivity, we have $L'' \leq_P L$.

Thus, L is NP-hard. □

Proving NP-completeness (2)

Method to prove that a language L is NP-complete:

- ① Prove $L \in \text{NP}$
- ② Prove L is NP-hard.
 - Select a known NP-complete language L' .
 - Describe an algorithm that computes a function f mapping every instance $x \in \Sigma^*$ of L' to an instance $f(x)$ of L .
 - Prove that $x \in L' \Leftrightarrow f(x) \in L$ for all $x \in \Sigma^*$.
 - Prove that the algorithm computing f runs in polynomial time.

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3-CNF SAT is NP-hard

Theorem 17

3-CNF SAT is NP-complete.

Proof.

3-CNF SAT is NP-hard

Theorem 17

3-CNF SAT is **NP**-complete.

Proof.

3-CNF SAT is in **NP**, since it is a special case of CNF-SAT.

3-CNF SAT is NP-hard

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Proof.

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To show that 3-CNF SAT is NP-hard, we give a polynomial reduction from CNF-SAT.

3-CNF SAT is NP-hard

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To show that 3-CNF SAT is NP-hard, we give a polynomial reduction from CNF-SAT.

Let F be a CNF formula. The reduction algorithm constructs a 3-CNF formula F' as follows. For each clause C in F :

- If C has at most 3 literals, then copy C into F' .
- Otherwise, denote $C = (\ell_1 \vee \ell_2 \vee \dots \vee \ell_k)$.

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- If C has at most 3 literals, then copy C into F' .
- Otherwise, denote $C = (\ell_1 \vee \ell_2 \vee \dots \vee \ell_k)$. Create $k - 3$ new variables y_1, \dots, y_{k-3} , and add the clauses $(\ell_1 \vee \ell_2 \vee y_1), (\neg y_1 \vee \ell_3 \vee y_2), (\neg y_2 \vee \ell_4 \vee y_3), \dots, (\neg y_{k-3} \vee \ell_{k-1} \vee \ell_k)$.

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Show that F is satisfiable $\Leftrightarrow F'$ is satisfiable.

Show that F' can be computed in polynomial time (trivial; use a RAM). □

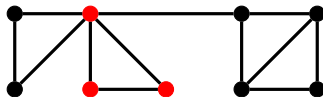
Clique

A **clique** in a graph $G = (V, E)$ is a subset of vertices $S \subseteq V$ such that every two vertices of S are adjacent in G .

CLIQUE

Input: Graph G , integer k

Question: Does G have a clique of size k ?



Theorem 18

CLIQUE is NP-complete.

Clique (2)

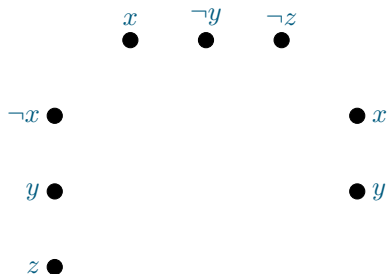
- CLIQUE is in NP

Clique (2)

- CLIQUE is in NP
- Let $F = C_1 \wedge C_2 \wedge \dots \wedge C_k$ be a 3-CNF formula
- Construct a graph G that has a clique of size k iff F is satisfiable

$$(\neg x \vee y \vee z) \wedge (x \vee \neg y \vee \neg z) \wedge (x \vee y)$$

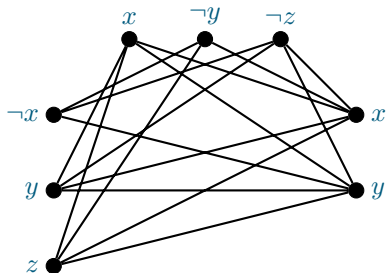
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- CLIQUE is in **NP**
- Let $F = C_1 \wedge C_2 \wedge \dots \wedge C_k$ be a 3-CNF formula
- Construct a graph G that has a clique of size k iff F is satisfiable
- For each clause $C_r = (\ell_1^r \vee \dots \vee \ell_w^r)$, $1 \leq r \leq k$, create w new vertices v_1^r, \dots, v_w^r

$$(\neg x \vee y \vee z) \wedge (x \vee \neg y \vee \neg z) \wedge (x \vee y)$$

Clique (2)



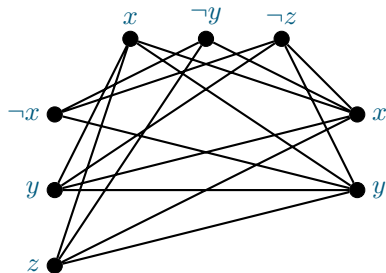
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- CLIQUE is in **NP**
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- Add an edge between v_i^r and v_j^s if

$$r \neq s \quad \text{and} \quad \ell_i^r \neq \neg \ell_j^s \quad \text{where } \neg \neg x = x.$$

- Check correctness and polynomial running time

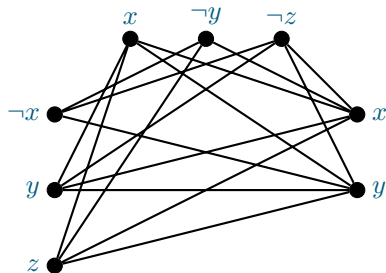
Clique (2)



- Correctness: F has a satisfying assignment iff G has a clique of size k .

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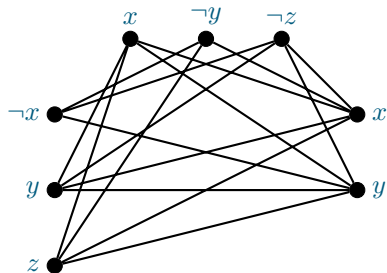
Clique (2)



- Correctness: F has a satisfying assignment iff G has a clique of size k .
- (\Rightarrow): Let α be a sat. assignment for F . For each clause C_r , choose a literal ℓ_i^r with $\alpha(\ell_i^r) = 1$, and denote by s^r the corresponding vertex in G . Now, $\{s^r : 1 \leq r \leq k\}$ is a clique of size k in G since $\alpha(x) \neq \alpha(\neg x)$.

$$(\neg x \vee y \vee z) \wedge (x \vee \neg y \vee \neg z) \wedge (x \vee y)$$

Clique (2)



$$(\neg x \vee y \vee z) \wedge (x \vee \neg y \vee \neg z) \wedge (x \vee y)$$

- Correctness: F has a satisfying assignment iff G has a clique of size k .
- (\Rightarrow) : Let α be a sat. assignment for F . For each clause C_r , choose a literal ℓ_i^r with $\alpha(\ell_i^r) = 1$, and denote by s^r the corresponding vertex in G . Now, $\{s^r : 1 \leq r \leq k\}$ is a clique of size k in G since $\alpha(x) \neq \alpha(\neg x)$.
- (\Leftarrow) : Let S be a clique of size k in G . Then, S contains exactly one vertex $s_r \in \{v_1^r, \dots, v_w^r\}$ for each $r \in \{1, \dots, k\}$. Denote by l^r the corresponding literal. Now, for any r, r' , it is not the case that $l_r = \neg l_{r'}$. Therefore, there is an assignment α to $\text{var}(F)$ such that $\alpha(l_r) = 1$ for each $r \in \{1, \dots, k\}$ and α satisfies F .

Vertex Cover

A **vertex cover** in a graph $G = (V, E)$ is a subset of vertices $S \subseteq V$ such that every edge of G has an endpoint in S .

VERTEX COVER

Input: Graph G , integer k

Question: Does G have a vertex cover of size k ?

Theorem 19

VERTEX COVER is **NP**-complete.

The proof is left as an exercise.

Hamiltonian Cycle

A **Hamiltonian Cycle** in a graph $G = (V, E)$ is a cycle visiting each vertex exactly once.

(Alternatively, a permutation of V such that every two consecutive vertices are adjacent and the first and last vertex in the permutation are adjacent.)

HAMILTONIAN CYCLE

Input: Graph G

Question: Does G have a Hamiltonian Cycle?

Theorem 20

HAMILTONIAN CYCLE is **NP**-complete.

Proof sketch.

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- HAMILTONIAN CYCLE is in **NP**: the certificate is a Hamiltonian Cycle of G .

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- HAMILTONIAN CYCLE is in **NP**: the certificate is a Hamiltonian Cycle of G .
- Let us show: VERTEX COVER \leq_P HAMILTONIAN CYCLE

...



Hamiltonian Cycle (2)

Theorem 21

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Let us show: VERTEX COVER \leq_P HAMILTONIAN CYCLE

Hamiltonian Cycle (2)

Theorem 21

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Let us show: VERTEX COVER \leq_P HAMILTONIAN CYCLE
- Let $(G = (V, E), k)$ be an instance for VERTEX COVER (VC).
- We will construct an equivalent instance G' for HAMILTONIAN CYCLE (HC).

Hamiltonian Cycle (2)

Theorem 21

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Let us show: VERTEX COVER \leq_P HAMILTONIAN CYCLE
- Let $(G = (V, E), k)$ be an instance for VERTEX COVER (VC).
- We will construct an equivalent instance G' for HAMILTONIAN CYCLE (HC).
- Intuition: Non-deterministic choices
 - for VC: which vertices to select in the vertex cover
 - for HC: which route the cycle takes

...



Hamiltonian Cycle (3)

Theorem 22

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Add k vertices s_1, \dots, s_k to G' (*selector vertices*)

Hamiltonian Cycle (3)

Theorem 22

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Add k vertices s_1, \dots, s_k to G' (*selector vertices*)
- Each edge of G will be represented by a gadget (subgraph) of G'
- s.t. the set of edges covered by a vertex x in G corresponds to a partial cycle going through all gadgets of G' representing these edges.

Hamiltonian Cycle (3)

Theorem 22

HAMILTONIAN CYCLE is NP-complete.

Proof sketch (continued).

- Add k vertices s_1, \dots, s_k to G' (*selector vertices*)
- Each edge of G will be represented by a gadget (subgraph) of G'
- s.t. the set of edges covered by a vertex x in G corresponds to a partial cycle going through all gadgets of G' representing these edges.
- Attention: we need to allow for an edge to be covered by both endpoints

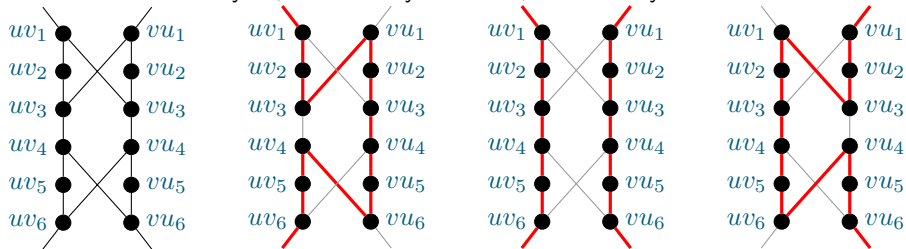
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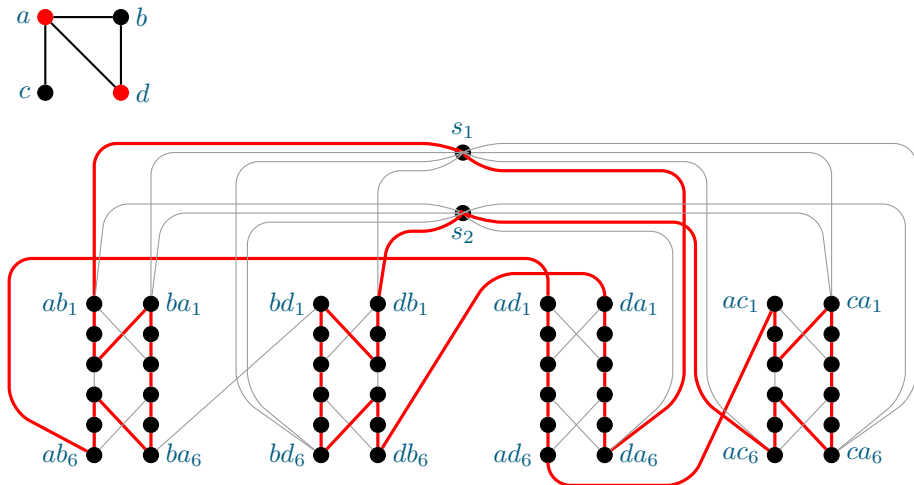
Hamiltonian Cycle (4)

Gadget representing the edge $\{u, v\} \in E$

Its states: 'covered by u ', 'covered by u and v ', 'covered by v '



Hamiltonian Cycle (5)



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Further Reading

- Chapter 34, **NP-Completeness**, in (Cormen et al., 2009)
- Garey and Johnson's influential reference book (Garey and Johnson, 1979)

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- Thomas H. Cormen, Charles E. Leiserson, Ronald L. Rivest, and Clifford Stein (2009). *Introduction to Algorithms*. 3rd ed. The MIT Press.
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Kernelization

Serge Gaspers

UNSW

- 1 Vertex Cover
 - Simplification rules
 - Preprocessing algorithm
- 2 Kernelization algorithms
- 3 Kernel for HAMILTONIAN CYCLE
- 4 Kernel for EDGE CLIQUE COVER
- 5 Kernels and Fixed-parameter tractability
- 6 Further Reading

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Vertex cover

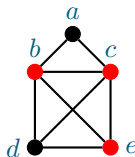
A **vertex cover** of a graph $G = (V, E)$ is a subset of vertices $S \subseteq V$ such that for each edge $\{u, v\} \in E$, we have $u \in S$ or $v \in S$.

VERTEX COVER

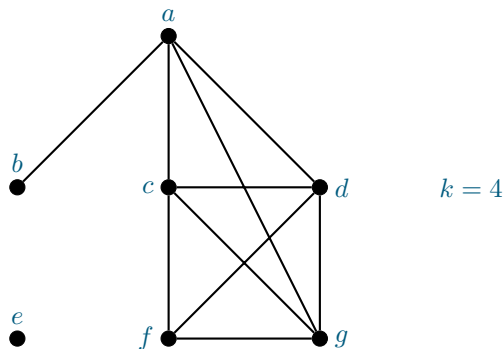
Input: A graph $G = (V, E)$ and an integer k

Parameter: k

Question: Does G have a vertex cover of size at most k ?



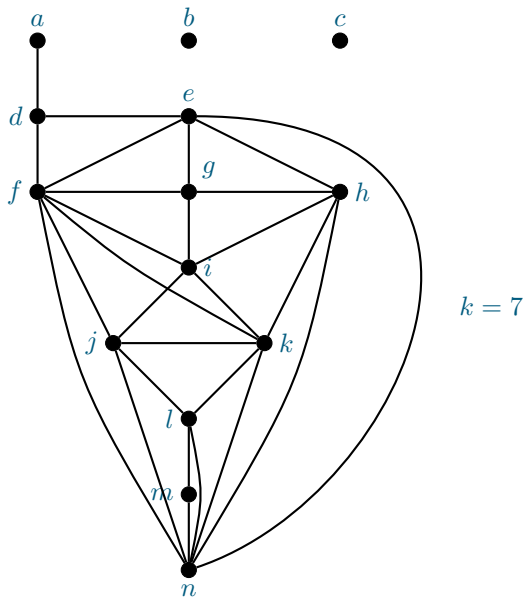
Exercise 1



Is this a **YES**-instance for VERTEX COVER?

(Is there $S \subseteq V$ with $|S| \leq 4$, such that $\forall uv \in E, u \in S$ or $v \in S$?)

Exercise 2



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Simplification rules for VERTEX COVER

(Degree-0)

If $\exists v \in V$ such that $d_G(v) = 0$, then set $G \leftarrow G - v$.

Simplification rules for VERTEX COVER

(Degree-0)

If $\exists v \in V$ such that $d_G(v) = 0$, then set $G \leftarrow G - v$.

Proving correctness. A simplification rule is **sound** if for every instance, it produces an equivalent instance. Two instances I, I' are **equivalent** if they are both **YES**-instances or they are both **NO**-instances.

Lemma 1

(Degree-0) is sound.

Simplification rules for VERTEX COVER

(Degree-0)

If $\exists v \in V$ such that $d_G(v) = 0$, then set $G \leftarrow G - v$.

Proving correctness. A simplification rule is **sound** if for every instance, it produces an equivalent instance. Two instances I, I' are **equivalent** if they are both **YES**-instances or they are both **NO**-instances.

Lemma 1

(Degree-0) is sound.

Proof.

First, suppose $(G - v, k)$ is a **YES**-instance. Let S be a vertex cover for $G - v$ of size at most k . Then, S is also a vertex cover for G since no edge of G is incident to v . Thus, (G, k) is a **YES**-instance.

Now, suppose $(G - v, k)$ is a **NO**-instance. For the sake of contradiction, assume (G, k) is a **YES**-instance. Let S be a vertex cover for G of size at most k . But then, $S \setminus \{v\}$ is a vertex cover of size at most k for $G - v$; a contradiction. \square

Simplification rules for VERTEX COVER

(Degree-1)

If $\exists v \in V$ such that $d_G(v) = 1$, then set $G \leftarrow G - N_G[v]$ and $k \leftarrow k - 1$.

Simplification rules for VERTEX COVER

(Degree-1)

If $\exists v \in V$ such that $d_G(v) = 1$, then set $G \leftarrow G - N_G[v]$ and $k \leftarrow k - 1$.

Lemma 1

(Degree-1) is sound.

Proof.

Let u be the neighbor of v in G . Thus, $N_G[v] = \{u, v\}$.

If S is a vertex cover of G of size at most k , then $S \setminus \{u, v\}$ is a vertex cover of $G - N_G[v]$ of size at most $k - 1$, because $u \in S$ or $v \in S$.

If S' is a vertex cover of $G - N_G[v]$ of size at most $k - 1$, then $S' \cup \{u\}$ is a vertex cover of G of size at most k , since all edges that are in G but not in $G - N_G[v]$ are incident to u . □

Simplification rules for VERTEX COVER

(Large Degree)

If $\exists v \in V$ such that $d_G(v) > k$, then set $G \leftarrow G - v$ and $k \leftarrow k - 1$.

Simplification rules for VERTEX COVER

(Large Degree)

If $\exists v \in V$ such that $d_G(v) > k$, then set $G \leftarrow G - v$ and $k \leftarrow k - 1$.

Lemma 1

(Large Degree) is sound.

Proof.

Let S be a vertex cover of G of size at most k . If $v \notin S$, then $N_G(v) \subseteq S$, contradicting that $|S| \leq k$. □

Simplification rules for VERTEX COVER

(Number of Edges)

If $d_G(v) \leq k$ for each $v \in V$ and $|E| > k^2$ then return No

Simplification rules for VERTEX COVER

(Number of Edges)

If $d_G(v) \leq k$ for each $v \in V$ and $|E| > k^2$ then return **No**

Lemma 1

(Number of Edges) is sound.

Proof.

Assume $d_G(v) \leq k$ for each $v \in V$ and $|E| > k^2$.

Suppose $S \subseteq V$, $|S| \leq k$, is a vertex cover of G .

We have that S covers at most k^2 edges.

However, $|E| \geq k^2 + 1$.

Thus, S is not a vertex cover of G . □

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Preprocessing algorithm for VERTEX COVER

VC-preprocess

Input: A graph G and an integer k .

Output: A graph G' and an integer k' such that G has a vertex cover of size at most k if and only if G' has a vertex cover of size at most k' .

$G' \leftarrow G$

$k' \leftarrow k$

repeat

 Execute simplification rules (Degree-0), (Degree-1), (Large Degree), and
 (Number of Edges) for (G', k')

until *no simplification rule applies*

return (G', k')

Effectiveness of preprocessing algorithms

- How effective is VC-preprocess?
- We would like to study preprocessing algorithms mathematically and quantify their effectiveness.

- Say that a preprocessing algorithm for a problem Π is **nice** if it runs in polynomial time and for each instance for Π , it returns an instance for Π that is strictly smaller.

First try

- Say that a preprocessing algorithm for a problem Π is **nice** if it runs in polynomial time and for each instance for Π , it returns an instance for Π that is strictly smaller.
- \rightarrow executing it a linear number of times reduces the instance to a single bit
- \rightarrow such an algorithm would solve Π in polynomial time
- For **NP**-hard problems this is not possible unless $P = NP$
- We need a different measure of effectiveness

Measuring the effectiveness of preprocessing algorithms

- We will measure the effectiveness in terms of the **parameter**
- How large is the resulting instance in terms of the parameter?

Effectiveness of VC-preprocess

Lemma 2

For any instance (G, k) for VERTEX COVER, VC-preprocess produces an equivalent instance (G', k') of size $O(k^2)$.

Proof.

Since all simplification rules are sound, $(G = (V, E), k)$ and $(G' = (V', E'), k')$ are equivalent.

By (Number of Edges), $|E'| \leq (k')^2 \leq k^2$.

By (Degree-0) and (Degree-1), each vertex in V' has degree at least 2 in G' .

Since $\sum_{v \in V'} d_{G'}(v) = 2|E'| \leq 2k^2$, this implies that $|V'| \leq k^2$.

Thus, $|V'| + |E'| \leq O(k^2)$. □

Outline

- 1 Vertex Cover
 - Simplification rules
 - Preprocessing algorithm
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- 6 Further Reading

Definition 3

A **kernelization** for a parameterized problem Π is a **polynomial time** algorithm, which, for any instance I of Π with parameter k , produces an **equivalent** instance I' of Π with parameter k' such that $|I'| \leq f(k)$ and $k' \leq f(k)$ for a computable function f .

We refer to the function f as the **size** of the kernel.

Note: We do not formally require that $k' \leq k$, but this will be the case for many kernelizations.

VC-preprocess is a quadratic kernelization

Theorem 4

VC-preprocess is a $O(k^2)$ kernelization for VERTEX COVER.

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HAMILTONIAN CYCLE I

A **Hamiltonian cycle** of G is a subgraph of G that is a cycle on $|V(G)|$ vertices.

vc-HAMILTONIAN CYCLE

Input: A graph $G = (V, E)$.

Parameter: $k = vc(G)$, the size of a smallest vertex cover of G .

Question: Does G have a Hamiltonian cycle?

Thought experiment: Imagine a very large instance where the parameter is tiny. How can you simplify such an instance?

Issue: We do not actually know a vertex cover of size k .
We do not even know the value of k (it is not part of the input).

HAMILTONIAN CYCLE III

- Obtain a vertex cover using an approximation algorithm. We will use a 2-approximation algorithm, producing a vertex cover of size $\leq 2k$ in polynomial time.
- If C is a vertex cover of size $\leq 2k$, then $I = V \setminus C$ is an independent set of size $\geq |V| - 2k$.
- No two consecutive vertices in the Hamiltonian Cycle can be in I .
- A kernel with $\leq 4k$ vertices can now be obtained with the following simplification rule.

(Too-large)

Compute a vertex cover C of size $\leq 2k$ in polynomial time.

If $2|C| < |V|$, then return **No**

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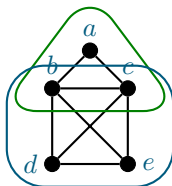
EDGE CLIQUE COVER

Definition 5

An **edge clique cover** of a graph $G = (V, E)$ is a set of cliques in G covering all its edges.

In other words, if $\mathcal{C} \subseteq 2^V$ is an edge clique cover then each $S \in \mathcal{C}$ is a clique in G and for each $\{u, v\} \in E$ there exists an $S \in \mathcal{C}$ such that $u, v \in S$.

Example: $\{\{a, b, c\}, \{b, c, d, e\}\}$ is an edge clique cover for this graph.



EDGE CLIQUE COVER

EDGE CLIQUE COVER

Input: A graph $G = (V, E)$ and an integer k

Parameter: k

Question: Does G have an edge clique cover of size at most k ?

The **size** of an edge clique cover \mathcal{C} is the number of cliques contained in \mathcal{C} and is denoted $|\mathcal{C}|$.

Helpful properties

Definition 5

A clique S in a graph G is a **maximal** clique if there is no other clique S' in G with $S \subset S'$.

Lemma 6

A graph G has an edge clique cover \mathcal{C} of size at most k if and only if G has an edge clique cover \mathcal{C}' of size at most k such that each $S \in \mathcal{C}'$ is a maximal clique.

Proof sketch.

(\Rightarrow): Replace each clique $S \in \mathcal{C}$ by a maximal clique S' with $S \subseteq S'$.

(\Leftarrow): Trivial, since \mathcal{C}' is an edge clique cover of size at most k . □

Simplification rules for EDGE CLIQUE COVER

Thought experiment: Imagine a very large instance where the parameter is tiny. How can you simplify such an instance?

Simplification rules for EDGE CLIQUE COVER II

The instance could have many degree-0 vertices.

(Isolated)

If there exists a vertex $v \in V$ with $d_G(v) = 0$, then set $G \leftarrow G - v$.

Lemma 7

(Isolated) is sound.

Proof sketch.

Since no edge is incident to v , a smallest edge clique cover for $G - v$ is a smallest edge clique cover for G , and vice-versa. \square

Simplification rules for EDGE CLIQUE COVER II

The instance could have many degree-0 vertices.

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If there exists a vertex $v \in V$ with $d_G(v) = 0$, then set $G \leftarrow G - v$.

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(Isolated) is sound.

Proof sketch.

Since no edge is incident to v , a smallest edge clique cover for $G - v$ is a smallest edge clique cover for G , and vice-versa. \square

(Isolated-Edge)

If $\exists uv \in E$ such that $d_G(u) = d_G(v) = 1$, then set $G \leftarrow G - \{u, v\}$ and $k \leftarrow k - 1$.

Simplification rules for EDGE CLIQUE COVER III

(Twins)

If $\exists u, v \in V$, $u \neq v$, such that $N_G[u] = N_G[v]$, then set $G \leftarrow G - v$.

Lemma 8

(Twins) is sound.

Simplification rules for EDGE CLIQUE COVER III

(Twins)

If $\exists u, v \in V$, $u \neq v$, such that $N_G[u] = N_G[v]$, then set $G \leftarrow G - v$.

Lemma 8

(Twins) is sound.

Proof.

We need to show that G has an edge clique cover of size at most k if and only if $G - v$ has an edge clique cover of size at most k .

(\Rightarrow): If \mathcal{C} is an edge clique cover of G of size at most k , then $\{S \setminus \{v\} : S \in \mathcal{C}\}$ is an edge clique cover of $G - v$ of size at most k .

(\Leftarrow): Let \mathcal{C}' be an edge clique cover of $G - v$ of size at most k . Partition \mathcal{C}' into $\mathcal{C}'_u = \{S \in \mathcal{C}' : u \in S\}$ and $\mathcal{C}'_{\neg u} = \mathcal{C}' \setminus \mathcal{C}'_u$. Note that each set in $\mathcal{C}_u = \{S \cup \{v\} : S \in \mathcal{C}'_u\}$ is a clique in G since $N_G[u] = N_G[v]$ and that each edge incident to v is contained in at least one of these cliques. Now, $\mathcal{C}_u \cup \mathcal{C}'_{\neg u}$ is an edge clique cover of G of size at most k . \square

Simplification rules for EDGE CLIQUE COVER IV

(Size-V)

If the previous simplification rules do not apply and $|V| > 2^k$, then return **No**.

Lemma 9

(Size-V) is sound.

Simplification rules for EDGE CLIQUE COVER IV

(Size-V)

If the previous simplification rules do not apply and $|V| > 2^k$, then return **No**.

Lemma 9

(Size-V) is sound.

Proof.

For the sake of contradiction, assume neither (Isolated) nor (Twins) are applicable, $|V| > 2^k$, and G has an edge clique cover \mathcal{C} of size at most k . Since $2^{\mathcal{C}}$ (the set of all subsets of \mathcal{C}) has size at most 2^k , and every vertex belongs to at least one clique in \mathcal{C} by (Isolated), we have that there exists two vertices $u, v \in V$ such that $\{S \in \mathcal{C} : u \in S\} = \{S \in \mathcal{C} : v \in S\}$. But then, $N_G[u] = \bigcup_{S \in \mathcal{C}: u \in S} S = \bigcup_{S \in \mathcal{C}: v \in S} S = N_G[v]$, contradicting that (Twin) is not applicable. \square

Kernel for EDGE CLIQUE COVER

Theorem 10 ((Gramm et al., 2008))

EDGE CLIQUE COVER *has a kernel with $O(2^k)$ vertices and $O(4^k)$ edges.*

Corollary 11

EDGE CLIQUE COVER *is FPT.*

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Theorem 12

Let Π be a decidable parameterized problem.

Π has a kernelization algorithm $\Leftrightarrow \Pi$ is FPT.

Kernels and Fixed-parameter tractability

Theorem 12

Let Π be a decidable parameterized problem.

Π has a kernelization algorithm $\Leftrightarrow \Pi$ is FPT.

Proof.

(\Rightarrow): An FPT algorithm is obtained by first running the kernelization, and then any brute-force algorithm on the resulting instance.

(\Leftarrow): Let A be an FPT algorithm for Π with running time $O(f(k)n^c)$.

If $f(k) < n$, then A has running time $O(n^{c+1})$. In this case, the kernelization algorithm runs A and returns a trivial YES- or NO-instance depending on the answer of A .

Otherwise, $f(k) \geq n$. In this case, the kernelization algorithm outputs the input instance. □

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Further Reading

- Chapter 2, *Kernelization* in (Cygan et al., 2015)
- Chapter 4, *Kernelization* in (Downey and Fellows, 2013)
- Chapter 7, *Data Reduction and Problem Kernels* in (Niedermeier, 2006)
- Chapter 9, *Kernelization and Linear Programming Techniques* in (Flum and Grohe, 2006)
- the kernelization book (Fomin et al., 2019)

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