# Computer Science & Engineering 423/823 Design and Analysis of Algorithms

Lecture 08 — NP-Completeness (Chapter 34)

Stephen Scott
(Adapted from Vinodchandran N. Variyam)



## Introduction





- So far, we have focused on problems with "efficient" algorithms
- I.e. problems with algorithms that run in polynomial time: $O(n^c)$ for some constant $c$
  - Side note: We call it efficient even if $c$ is large, since it is likely that another, even more efficient, algorithm exists
- But, for some problems, the fastest known algorithms require time that is **superpolynomial**
  - Includes sub-exponential time (e.g. $2^{n^{1/3}}$), exponential time (e.g. $2^n$), doubly exponential time (e.g. $2^{2^n}$), etc.
  - There are even problems that cannot be solved in *any* amount of time (e.g. the "halting problem")



## P vs. NP





- Our focus will be on the **complexity classes** called P and NP
- Centers on the notion of a **Turing machine** (TM), which is a finite state machine with an infinitely long tape for storage
  - Anything a computer can do, a TM can do, and vice-versa
  - More on this in CSCE 428/828 and CSCE 424/824
- P = "deterministic polynomial time" = the set of problems that can be solved by a deterministic TM (deterministic algorithm) in polynomial time
- NP = "nondeterministic polynomial time" = the set of problems that can be solved by a nondeterministic TM in polynomial time
  - Can loosely think of a nondeterministic TM as one that can explore many, many possible paths of computation at once
  - Equivalently, NP is the set of problems whose solutions, if given, can be **verified** in polynomial time



## P vs. NP Example





- Problem HAM-CYCLE: Does a graph $G = (V, E)$ contain a **hamiltonian cycle**, i.e. a simple cycle that visits every vertex in $V$ exactly once?
  - This problem is in NP, since if we were given a specific $G$ plus the answer to the question plus a **certificate**, we can verify a "yes" answer in polynomial time using the certificate
  - What would be an appropriate certificate?
  - Not known if HAM-CYCLE $\in$ P



## P vs. NP Example (2)





- Problem EULER: Does a directed graph $G = (V, E)$ contain an **Euler tour**, i.e. a cycle that visits every edge in $E$ exactly once and can visit vertices multiple times?
  - This problem is in P, since we can answer the question in polynomial time by checking if each vertex's in-degree equals its out-degree
  - Does that mean that the problem is also in NP? If so, what is the certificate?



## NP-Completeness





- Any problem in P is also in NP, since if we can efficently solve the problem, we get the poly-time verification for free
  - $\Rightarrow$ P $\subseteq$ NP
- Not known if P $\subset$ NP, i.e. unknown if there a problem in NP that's not in P
- A subset of the problems in NP is the set of **NP-complete** (NPC) problems
  - Every problem in NPC is at least as hard as all others in NP
  - These problems are believed to be intractable (no efficient algorithm), but not yet proven to be so
  - If any NPC problem is in P, then P = NP and life is glorious ⌣̈

## Proving NP-Completeness

- Thus, if we prove that a problem is NPC, we can tell our boss that we cannot find an efficient algorithm and should take a different approach
  - E.g. approximation algorithm, heuristic approach
- How do we prove that a problem $A$ is NPC?
  1. Prove that $A \in$ NP by finding certificate
  2. Show that $A$ is as hard as any other NP problem by showing that if we can efficiently solve $A$ then we can efficiently solve all problems in NP
- First step is usually easy, but second looks difficult
- Fortunately, part of the work has been done for us ...





## Reductions

- We will use the idea of a **reduction** of one problem to another to prove how hard it is
- A reduction takes an instance of one problem $A$ and transforms it to an instance of another problem $B$ in such a way that a solution to the instance of $B$ yields a solution to the instance of $A$
- Example 1: How did we solve the bipartite matching problem?
- Example 2: How did we solve the topological sort problem?
- Time complexity of reduction-based algorithm for $A$ is the time for the reduction to $B$ plus the time to solve the instance of $B$





## Decision Problems

- Before we go further into reductions, we simplify our lives by focusing on **decision problems**
- In a decision problem, the only output of an algorithm is an answer "yes" or "no"
- I.e. we're not asked for a shortest path or a hamiltonian cycle, etc.
- Not as restrictive as it may seem: Rather than asking for the weight of a shortest path from $i$ to $j$, just ask if there exists a path from $i$ to $j$ with weight at most $k$
- Such decision versions of *optimization problems* are no harder than the original optimization problem (why?), so if we show the decision version is hard, then so is the optimization version
- Decision versions are especially convenient when thinking in terms of languages and the Turing machines that accept/reject them





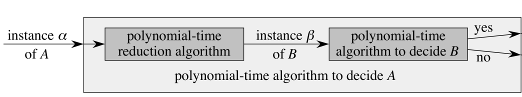
## Reductions (2)

- What is a reduction in the NPC sense?
- Start with two problems $A$ and $B$, and we want to show that problem $B$ is at least as hard as $A$
- Will **reduce** $A$ to $B$ via a **polynomial-time reduction** by transforming *any* instance $\alpha$ of $A$ to *some* instance $\beta$ of $B$ such that
  1. The transformation must take polynomial time (since we're talking about hardness in the sense of efficient vs. inefficient algorithms)
  2. The answer for $\alpha$ is "yes" if and only if the answer for $\beta$ is "yes"
- If such a reduction exists, then $B$ is at least as hard as $A$ since if an efficient algorithm exists for $B$, we can solve any instance of $A$ in polynomial time
- Notation: $A \leq_P B$, which reads as "$A$ is no harder to solve than $B$, modulo polynomial time reductions"





## Reductions (3)








## Reductions (4)

- But if we want to prove that a problem $B$ is NPC, do we have to reduce to it *every* problem in NP?
- No we don't:
  - If another problem $A$ is known to be NPC, then we know that any problem in NP reduces to it
  - If we reduce $A$ to $B$, then any problem in NP can reduce to $B$ via its reduction to $A$ followed by $A$'s reduction to $B$
  - We then can call $B$ an **NP-hard** problem, which is NPC if it is also in NP
  - Still need our first NPC problem to use as a basis for our reductions

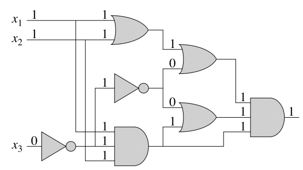
## CIRCUIT-SAT

- Our first NPC problem: CIRCUIT-SAT
- An instance is a boolean combinational circuit (no feedback, no memory)
- Question: Is there a **satisfying assignment**, i.e. an assignment of inputs to the circuit that satisfies it (makes its output 1)?

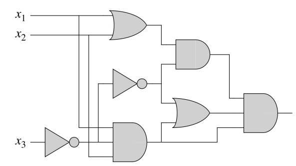




## CIRCUIT-SAT (2)



Satisfiable



Unsatisfiable





## CIRCUIT-SAT (3)

- To prove CIRCUIT-SAT to be NPC, need to show:
  1. CIRCUIT-SAT $\in$ NP; what is its certificate that we can confirm in polynomial time?
  2. That any problem in NP reduces to CIRCUIT-SAT
- We'll skip the NP-hardness proof, save to say that it leverages the existence of an algorithm that verifies certificates for some NP problem





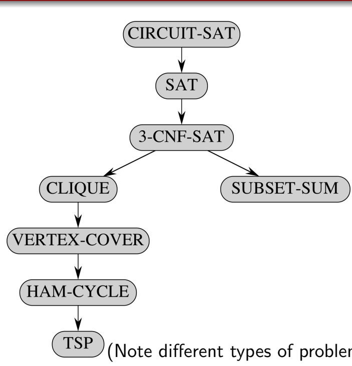
## Other NPC Problems

- We'll use the fact that CIRCUIT-SAT is NPC to prove that these other problems are as well:
  - SAT: Does boolean formula $\phi$ have a satisfying assignment?
  - 3-CNF-SAT: Does 3-CNF formula $\phi$ have a satisfying assignment?
  - CLIQUE: Does graph $G$ have a clique (complete subgraph) of $k$ vertices?
  - VERTEX-COVER: Does graph $G$ have a vertex cover (set of vertices that touches all edges) of $k$ vertices?
  - HAM-CYCLE: Does graph $G$ have a hamiltonian cycle?
  - TSP: Does complete, weighted graph $G$ have a hamiltonian cycle of total weight $\leq k$?
  - SUBSET-SUM: Is there a subset $S'$ of finite set $S$ of integers that sum to exactly a specific target value $t$?
- Many more in Garey & Johnson's book, with proofs





## Other NPC Problems (2)



(Note different types of problems involved in reductions)





## NPC Problem: Formula Satisfiability (SAT)

- Given: A boolean formula $\phi$ consisting of
  1. $n$ boolean variables $x_1, \ldots, x_n$
  2. $m$ boolean connectives from $\wedge$, $\vee$, $\neg$, $\rightarrow$, and $\leftrightarrow$
  3. Parentheses
- Question: Is there an assignment of boolean values to $x_1, \ldots, x_n$ to make $\phi$ evaluate to 1?
- E.g.: $\phi = ((x_1 \rightarrow x_2) \vee \neg((\neg x_1 \leftrightarrow x_3) \vee x_4)) \wedge \neg x_2$ has satisfying assignment $x_1 = 0$, $x_2 = 0$, $x_3 = 1$, $x_4 = 1$ since

$$\begin{aligned}\phi &= ((0 \rightarrow 0) \vee \neg((\neg 0 \leftrightarrow 1) \vee 1)) \wedge \neg 0 \\ &= (1 \vee \neg((1 \leftrightarrow 1) \vee 1)) \wedge 1 \\ &= (1 \vee \neg(1 \vee 1)) \wedge 1 \\ &= (1 \vee 0) \wedge 1 \\ &= 1\end{aligned}$$

## SAT is NPC

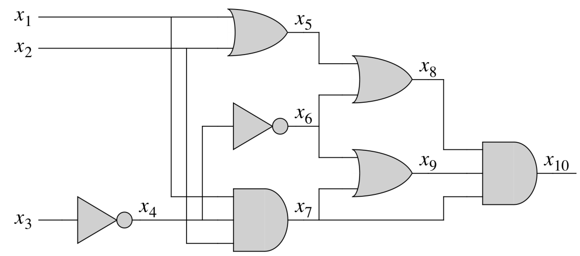
- SAT is in NP: $\phi$'s satisfying assignment certifies that the answer is "yes" and this can be easily checked in poly time
- SAT is NP-hard: Will show CIRCUIT-SAT $\leq_P$ SAT by reducing from CIRCUIT-SAT to SAT
- In reduction, need to map *any* instance (circuit) $C$ of CIRCUIT-SAT to *some* instance (formula) $\phi$ of SAT such that $C$ has a satisfying assignment if and only if $\phi$ does
- Further, the time to do the mapping must be polynomial in the size of the circuit, implying that $\phi$'s representation must be polynomially sized



Nebraska Lincoln
CSCE423/823
Introduction
Proofs of NPC Problems
SAT
3-CNF-SAT
CLIQUE
VERTEX-COVER
SUBSET-SUM

## SAT is NPC (2)

Define a variable in $\phi$ for each wire in $C$:






Nebraska Lincoln
CSCE423/823
Introduction
Proofs of NPC Problems
SAT
3-CNF-SAT
CLIQUE
VERTEX-COVER
SUBSET-SUM

## SAT is NPC (3)

- Then define a clause of $\phi$ for each gate that defines the function for that gate:

$$\begin{aligned}\phi = x_{10} \quad & \wedge \quad (x_4 \leftrightarrow \neg x_3) \\ & \wedge \quad (x_5 \leftrightarrow (x_1 \vee x_2)) \\ & \wedge \quad (x_6 \leftrightarrow \neg x_4) \\ & \wedge \quad (x_7 \leftrightarrow (x_1 \wedge x_2 \wedge x_4)) \\ & \wedge \quad (x_8 \leftrightarrow (x_5 \vee x_6)) \\ & \wedge \quad (x_9 \leftrightarrow (x_6 \vee x_7)) \\ & \wedge \quad (x_{10} \leftrightarrow (x_7 \wedge x_8 \wedge x_9))\end{aligned}$$



Nebraska Lincoln
CSCE423/823
Introduction
Proofs of NPC Problems
SAT
3-CNF-SAT
CLIQUE
VERTEX-COVER
SUBSET-SUM

## SAT is NPC (4)

- Size of $\phi$ is polynomial in size of $C$ (number of gates and wires)

$\Rightarrow$ If $C$ has a satisfying assignment, then the final output of the circuit is 1 and the value on each internal wire matches the output of the gate that feeds it
  - Thus, $\phi$ evaluates to 1

$\Leftarrow$ If $\phi$ has a satisfying assignment, then each of $\phi$'s clauses is satisfied, which means that each of $C$'s gate's output matches its function applied to its inputs, and the final output is 1

- Since satisfying assignment for $C \Rightarrow$ satisfying assignment for $\phi$ and vice-versa, we get $C$ has a satisfying assignment if and only if $\phi$ does



Nebraska Lincoln
CSCE423/823
Introduction
Proofs of NPC Problems
SAT
3-CNF-SAT
CLIQUE
VERTEX-COVER
SUBSET-SUM

## NPC Problem: 3-CNF Satisfiability (3-CNF-SAT)

- Given: A boolean formula that is in 3-conjunctive normal form (3-CNF), which is a conjunction of clauses, each a disjunction of 3 literals, e.g.

$$(x_1 \vee \neg x_1 \vee \neg x_2) \wedge (x_3 \vee x_2 \vee x_4) \wedge (\neg x_1 \vee \neg x_3 \vee \neg x_4) \wedge (x_4 \vee x_5 \vee x_1)$$

- Question: Is there an assignment of boolean values to $x_1, \ldots, x_n$ to make the formula evaluate to 1?



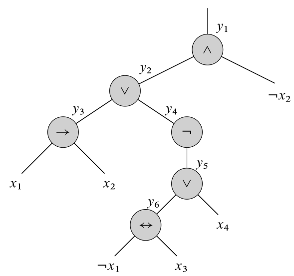
Nebraska Lincoln
CSCE423/823
Introduction
Proofs of NPC Problems
SAT
3-CNF-SAT
CLIQUE
VERTEX-COVER
SUBSET-SUM

## 3-CNF-SAT is NPC

- 3-CNF-SAT is in NP: The satisfying assignment certifies that the answer is "yes" and this can be easily checked in poly time
- 3-CNF-SAT is NP-hard: Will show SAT $\leq_P$ 3-CNF-SAT
- Again, need to map *any* instance $\phi$ of SAT to *some* instance $\phi'''$ of 3-CNF-SAT
  1. Parenthesize $\phi$ and build its *parse tree*, which can be viewed as a circuit
  2. Assign variables to wires in this circuit, as with previous reduction, yielding $\phi'$, a conjunction of clauses
  3. Use the truth table of each clause $\phi'_i$ to get its DNF, then convert it to CNF $\phi''_i$
  4. Add auxillary variables to each $\phi''_i$ to get three literals in it, yielding $\phi'''_i$
  5. Final CNF formula is $\phi''' = \bigwedge_i \phi'''_i$

## Building the Parse Tree

$\phi = ((x_1 \to x_2) \vee \neg((\neg x_1 \leftrightarrow x_3) \vee x_4)) \wedge \neg x_2$

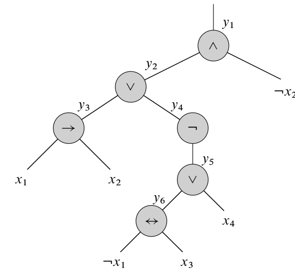



Might need to parenthesize $\phi$ to put at most two children per node





## Assign Variables to wires




$$\phi' = y_1 \wedge (y_1 \leftrightarrow (y_2 \wedge \neg x_2)) \wedge (y_2 \leftrightarrow (y_3 \vee y_4)) \wedge$$
$$(y_3 \leftrightarrow (x_1 \to x_2)) \wedge (y_4 \leftrightarrow \neg y_5) \wedge (y_5 \leftrightarrow (y_6 \vee x_4)) \wedge (y_6 \leftrightarrow (\neg x_1 \leftrightarrow x_3))$$





## Convert Each Clause to CNF

- Consider first clause $\phi'_1 = (y_1 \leftrightarrow (y_2 \wedge \neg x_2))$
- Truth table:

| $y_1$ | $y_2$ | $x_2$ | $(y_1 \leftrightarrow (y_2 \wedge \neg x_2))$ |
|---|---|---|---|
| 1 | 1 | 1 | 0 |
| 1 | 1 | 0 | 1 |
| 1 | 0 | 1 | 0 |
| 1 | 0 | 0 | 0 |
| 0 | 1 | 1 | 1 |
| 0 | 1 | 0 | 0 |
| 0 | 0 | 1 | 1 |
| 0 | 0 | 0 | 1 |

- Can now directly read off DNF of negation:

$$\neg\phi'_1 = (y_1 \wedge y_2 \wedge x_2) \vee (y_1 \wedge \neg y_2 \wedge x_2) \vee (y_1 \wedge \neg y_2 \wedge \neg x_2) \vee (\neg y_1 \wedge y_2 \wedge \neg x_2)$$

- And use DeMorgan's Law to convert it to CNF:

$$\phi''_1 = (\neg y_1 \vee \neg y_2 \vee \neg x_2) \wedge (\neg y_1 \vee y_2 \vee \neg x_2) \wedge (\neg y_1 \vee y_2 \vee x_2) \wedge (y_1 \vee \neg y_2 \vee x_2)$$





## Add Auxillary Variables

- Based on our construction, $\phi = \phi'' = \bigwedge_i \phi''_i$, where each $\phi''_i$ is a CNF formula each with at most three literals per clause
- But we need to have *exactly* three per clause!
- Simple fix: For each clause $C_i$ of $\phi''$,
  1. If $C_i$ has three distinct literals, add it as a clause in $\phi'''$
  2. If $C_i = (\ell_1 \vee \ell_2)$ for distinct literals $\ell_1$ and $\ell_2$, then add to $\phi'''$ $(\ell_1 \vee \ell_2 \vee p) \wedge (\ell_1 \vee \ell_2 \vee \neg p)$
  3. If $C_i = (\ell)$, then add to $\phi'''$ $(\ell \vee p \vee q) \wedge (\ell \vee p \vee \neg q) \wedge (\ell \vee \neg p \vee q) \wedge (\ell \vee \neg p \vee \neg q)$
- $p$ and $q$ are **auxillary variables**, and the combinations in which they're added result in a logically equivalent expression to that of the original clause, regardless of the values of $p$ and $q$





## Proof of Correctness of Reduction

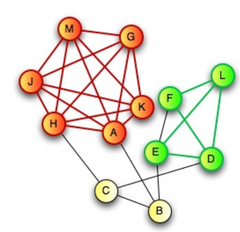
- $\phi$ has a satisfying assignment iff $\phi'''$ does
  1. CIRCUIT-SAT reduction to SAT implies satisfiability preserved from $\phi$ to $\phi'$
  2. Use of truth tables and DeMorgan's Law ensures $\phi''$ equivalent to $\phi'$
  3. Addition of auxillary variables ensures $\phi'''$ equivalent to $\phi''$
- Constructing $\phi'''$ from $\phi$ takes polynomial time
  1. $\phi'$ gets variables from $\phi$, plus at most one variable and one clause per operator in $\phi$
  2. Each clause in $\phi'$ has at most 3 variables, so each truth table has at most 8 rows, so each clause in $\phi'$ yields at most 8 clauses in $\phi''$
  3. Since there are only two auxillary variables, each clause in $\phi''$ yields at most 4 in $\phi'''$
  4. Thus size of $\phi'''$ is polynomial in size of $\phi$, and each step easily done in polynomial time





## NPC Problem: Clique Finding (CLIQUE)

- Given: An undirected graph $G = (V, E)$ and value $k$
- Question: Does $G$ contain a clique (complete subgraph) of size $k$?




Has a clique of size $k = 6$, but not of size 7

## CLIQUE is NPC



- CLIQUE is in NP: A list of vertices in the clique certifies that the answer is "yes" and this can be easily checked in poly time (how?)
- CLIQUE is NP-hard: Will show 3-CNF-SAT $\leq_P$ CLIQUE by mapping *any* instance $\phi$ of 3-CNF-SAT to *some* instance $\langle G, k \rangle$ of CLIQUE
  - Seems strange to reduce a boolean formula to a graph, but we will show that $\phi$ has a satisfying assignment iff $G$ has a clique of size $k$
  - Caveat: the reduction merely preserves the iff relationship; it does not try to directly solve either problem, nor does it assume it knows what the answer is



## The Reduction



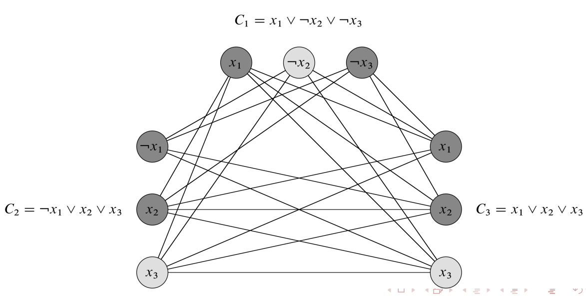
- Let $\phi = C_1 \wedge \cdots \wedge C_k$ be a 3-CNF formula with $k$ clauses
- For each clause $C_r = (\ell_1^r \vee \ell_2^r \vee \ell_3^r)$ put vertices $v_1^r$, $v_2^r$, and $v_3^r$ into $V$
- Add edge $(v_i^r, v_j^s)$ to $E$ if:
  1. $r \neq s$, i.e. $v_i^r$ and $v_j^s$ are in separate triples
  2. $\ell_i^r$ is not the negation of $\ell_j^s$
- Obviously can be done in polynomial time



## The Reduction (2)



$\phi = (x_1 \vee \neg x_2 \vee \neg x_3) \wedge (\neg x_1 \vee x_2 \vee x_3) \wedge (x_1 \vee x_2 \vee x_3)$
Satisfied by $x_2 = 0$, $x_3 = 1$






## The Reduction (3)



$\Rightarrow$ If $\phi$ has a satisfying assignment, then at least one literal in each clause is true
- Picking corresponding vertex from a true literal from each clause yields a set $V'$ of $k$ vertices, each in a distinct triple
- Since each vertex in $V'$ is in a distinct triple and literals that are negations of each other cannot both be true in a satisfying assignment, there is an edge between each pair of vertices in $V'$
- $V'$ is a clique of size $k$

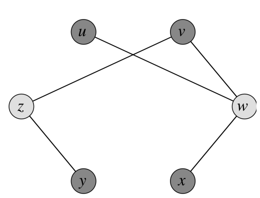
$\Leftarrow$ If $G$ has a size-$k$ clique $V'$, can assign 1 to corresponding literal of each vertex in $V'$
- Each vertex in its own triple, so each clause has a literal set to 1
- Will not try to set both a literal and its negation to 1
- Get a satisfying assignment



## NPC Problem: Vertex Cover Finding (VERTEX-COVER)



- A vertex in a graph is said to **cover** all edges incident to it
- A **vertex cover** of a graph is a set of vertices that covers all edges in the graph
- Given: An undirected graph $G = (V, E)$ and value $k$
- Question: Does $G$ contain a vertex cover of size $k$?




Has a vertex cover of size $k = 2$, but not of size 1



## VERTEX-COVER is NPC



- VERTEX-COVER is in NP: A list of vertices in the vertex cover certifies that the answer is "yes" and this can be easily checked in poly time
- VERTEX-COVER is NP-hard: Will show CLIQUE $\leq_P$ VERTEX-COVER by mapping *any* instance $\langle G, k \rangle$ of CLIQUE to *some* instance $\langle G', k' \rangle$ of VERTEX-COVER
- Reduction is simple: Given instance $\langle G = (V, E), k \rangle$ of CLIQUE, instance of VERTEX-COVER is $\langle \overline{G}, |V| - k \rangle$, where $\overline{G} = (V, \overline{E})$ is $G$'s **complement**:

$$\overline{E} = \{(u, v) : u, v \in V, u \neq v, (u, v) \notin E\}$$

- Easily done in polynomial time

## Proof of Correctness

- $\Rightarrow$ Assume $G$ has a size-$k$ clique $V' \subseteq V$
- Consider edge $(u, v) \in \overline{E}$
- If it's in $\overline{E}$, then $(u, v) \notin E$, so at least one of $u$ and $v$ (which cover $(u, v)$) is not in $V'$, so at least one of them is in $V \setminus V'$
- This holds for each edge in $\overline{E}$, so $V \setminus V'$ is a vertex cover of $\overline{G}$ of size $|V| - k$
- $\Leftarrow$ Assume $\overline{G}$ has a size-$(|V| - k)$ vertex cover $V'$
- For each $(u, v) \in \overline{E}$, at least one of $u$ and $v$ is in $V'$
- By contrapositive, if $u, v \notin V'$, then $(u, v) \in E$
- Since every pair of nodes in $V \setminus V'$ has an edge between them, $V \setminus V'$ is a clique of size $|V| - |V'| = k$





## NPC Problem: Subset Sum (SUBSET-SUM)

- Given: A finite set $S$ of positive integers and a positive integer **target** $t$
- Question: Is there a subset $S' \subseteq S$ whose elements sum to $t$?
- E.g. $S = \{1, 2, 7, 14, 49, 98, 343, 686, 2409, 2793, 16808, 17206, 117705, 117993\}$ and $t = 138457$ has a solution $S' = \{1, 2, 7, 98, 343, 686, 2409, 17206, 117705\}$





## SUBSET-SUM is NPC

- SUBSET-SUM is in NP: The subset $S'$ certifies that the answer is "yes" and this can be easily checked in poly time
- SUBSET-SUM is NP-hard: Will show 3-CNF-SAT $\leq_P$ SUBSET-SUM by mapping *any* instance $\phi$ of 3-CNF-SAT to *some* instance $\langle S, t \rangle$ of SUBSET-SUM
- Make two reasonable assumptions about $\phi$:
  1. No clause contains both a variable and its negation
  2. Each variable appears in at least one clause





## The Reduction

- Let $\phi$ have $k$ clauses $C_1, \ldots, C_k$ over $n$ variables $x_1, \ldots, x_n$
- Reduction creates two numbers in $S$ for each variable $x_i$ and two numbers for each clause $C_j$
- Each number has $n + k$ digits, the most significant $n$ tied to variables and least significant $k$ tied to clauses
  1. Target $t$ has a 1 in each digit tied to a variable and a 4 in each digit tied to a clause
  2. For each $x_i$, $S$ contains integers $v_i$ and $v_i'$, each with a 1 in $x_i$'s digit and 0 for other variables. Put a 1 in $C_j$'s digit for $v_i$ if $x_i$ in $C_j$, and a 1 in $C_j$'s digit for $v_i'$ if $\neg x_i$ in $C_j$
  3. For each $C_j$, $S$ contains integers $s_j$ and $s_j'$, where $s_j$ has a 1 in $C_j$'s digit and 0 elsewhere, and $s_j'$ has a 2 in $C_j$'s digit and 0 elsewhere
- Greatest sum of any digit is 6, so no carries when summing integers
- Can be done in polynomial time





## The Reduction (2)

$$C_1 = (x_1 \vee \neg x_2 \vee \neg x_3),\ C_2 = (\neg x_1 \vee \neg x_2 \vee \neg x_3),$$
$$C_3 = (\neg x_1 \vee \neg x_2 \vee x_3),\ C_4 = (x_1 \vee x_2 \vee x_3)$$

| | | $x_1$ | $x_2$ | $x_3$ | $C_1$ | $C_2$ | $C_3$ | $C_4$ |
|---|---|---|---|---|---|---|---|---|
| $v_1$ | = | 1 | 0 | 0 | 1 | 0 | 0 | 1 |
| $v_1'$ | = | 1 | 0 | 0 | 0 | 1 | 1 | 0 |
| $v_2$ | = | 0 | 1 | 0 | 0 | 0 | 0 | 1 |
| $v_2'$ | = | 0 | 1 | 0 | 1 | 1 | 1 | 0 |
| $v_3$ | = | 0 | 0 | 1 | 0 | 0 | 1 | 1 |
| $v_3'$ | = | 0 | 0 | 1 | 1 | 1 | 0 | 0 |
| $s_1$ | = | 0 | 0 | 0 | 1 | 0 | 0 | 0 |
| $s_1'$ | = | 0 | 0 | 0 | 2 | 0 | 0 | 0 |
| $s_2$ | = | 0 | 0 | 0 | 0 | 1 | 0 | 0 |
| $s_2'$ | = | 0 | 0 | 0 | 0 | 2 | 0 | 0 |
| $s_3$ | = | 0 | 0 | 0 | 0 | 0 | 1 | 0 |
| $s_3'$ | = | 0 | 0 | 0 | 0 | 0 | 2 | 0 |
| $s_4$ | = | 0 | 0 | 0 | 0 | 0 | 0 | 1 |
| $s_4'$ | = | 0 | 0 | 0 | 0 | 0 | 0 | 2 |
| $t$ | = | 1 | 1 | 1 | 4 | 4 | 4 | 4 |

$x_1 = 0,\ x_2 = 0,\ x_3 = 1$





## Proof of Correctness

- $\Rightarrow$ If $x_i = 1$ in $\phi$'s satisfying assignment, SUBSET-SUM solution $S'$ will have $v_i$, otherwise $v_i'$
- For each variable-based digit, the sum of the elements of $S'$ is 1
- Since each clause is satisfied, each clause contains at least one literal with the value 1, so each clause-based digit sums to 1, 2, or 3
- To match each clause-based digit in $t$, add in the appropriate subset of **slack variables** $s_i$ and $s_i'$

## Proof of Correctness (2)



$\Leftarrow$ In SUBSET-SUM solution $S'$, for each $i = 1, \ldots, n$, exactly one of $v_i$ and $v_i'$ must be in $S'$, or sum won't match $t$

- If $v_i \in S'$, set $x_i = 1$ in satisfying assignment, otherwise we have $v_i' \in S'$ and set $x_i = 0$
- To get a sum of 4 in clause-based digit $C_j$, $S'$ must include a $v_i$ or $v_i'$ value that is 1 in that digit (since slack variables sum to at most 3)
- Thus, if $v_i \in S'$ has a 1 in $C_j$'s position, then $x_i$ is in $C_j$ and we set $x_i = 1$, so $C_j$ is satisfied (similar argument for $v_i' \in S'$ and setting $x_i = 0$)
- This holds for all clauses, so $\phi$ is satisfied



## In-Class Exercise



- OK, everything perfectly clear?
- Want a shot at extra credit?
- Put away your books (keep your notes), split into groups, and get ready for an in-class exercise!