Computational Complexity Theory

Lecture 18: Set lower bound protocol;
Complexity of Counting:
class #P; #P-completeness

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Recap: Randomized reduction

- Definition. We say a L_1 reduces to a L_2 in *randomized polynomial-time*, denoted $L_1 \leq_r L_2$, if there's a poly-time PTM M s.t. for every $x \in \{0,1\}^*$

$$\Pr [L_1(x) = L_2(M(x))] \geq 2/3.$$

- For arbitrary L_1 and L_2, we may not be able to boost the success probability 2/3, and so, the above kind of reductions **needn't be transitive**. However,
- Obs. If $L_1 \leq_r L_2$ and $L_2 \in$ BPP, then $L_1 \in$ BPP.

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- Obs. If $L_2 = SAT$, then we can boost the success probability from $\frac{1}{2} + |x|^{-c}$ to $1 - \exp(-|x|^d)$.
- *Proof idea.* BPP error reduction trick + Cook-Levin.

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- Obs. If $L_2 = SAT$, then we can boost the success probability from $\frac{1}{2} + |x|^{-c}$ to $1 - \exp(-|x|^d)$.

- Recall, $NP = \{L : L \leq_p SAT\}$. It makes sense to define a similar class using randomized poly-time reduction.

Recap: Class BP.NP

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- Obs. If $L_2 = SAT$, then we can boost the success probability from $\frac{1}{2} + |x|^{-c}$ to $1 - \exp(-|x|^d)$.
- Definition. $BP.NP = \{L : L \leq_r SAT\}$.
- Class BP.NP is also known as AM (*Arthur-Merlin protocol*) in the literature.

Recap: Class BP.NP

- Definition. BP.NP = $\{L : L \leq_r SAT\}$.
- Observe that NP $\subseteq$ BP.NP and BPP $\subseteq$ BP.NP. Is BP.NP = NP ? Many believe that the answer is "yes".

- Theorem. If certain reasonable circuit lower bounds hold, then BP.NP = NP.
- *Proof idea*. Similar to Nisan & Wigderson's conditional BPP = P result.

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- We may further ask:

1. Is BP.NP in PH? Recall, BPP is in PH.

2. Is $\overline{SAT} \in$ BP.NP? Recall, if SAT $\in$ BPP then PH collapses. (SAT $\in$ BP.NP as NP $\subseteq$ BP.NP .)

Recap: Class BP.NP

- Definition. BP.NP = $\{L : L \leq_r SAT\}$.

- Theorem. BP.NP is in $\sum_3$. (In fact, BP.NP is in $\prod_2$.)
- *Proof idea.* Similar to the Sipser-Gacs-Lautemann theorem.

- Wondering if BP.NP $\subseteq \prod_2$ implies BP.NP $\subseteq \sum_2$? Is BP.NP = co-BP.NP ? (Recall, BPP = co-BPP).
- If BP.NP = co-BP.NP then co-NP $\subseteq$ BP.NP. The next theorem shows that this would lead to PH collapse.

Recap: Class BP.NP

- Definition. BP.NP = $\{L : L \leq_r SAT\}$.

- Theorem. If $\overline{SAT} \in$ BP.NP then PH = $\sum_3$ (in fact, PH = $\sum_2$).
- *Proof idea.* Similar to Adleman's theorem + Karp-Lipton theorem.

Recap: Class BP.NP

- Definition. BP.NP = $\{L : L \leq_r SAT\}$.
- Theorem. If $\overline{SAT} \in$ BP.NP then PH = $\sum_2$.
- We would use the above theorem to show that if GI is NP-complete then PH collapses.
- Thus, even without designing an efficient algorithm for GI, we know GI is unlikely to be NP-complete!

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- Theorem. *(Goldwasser-Sipser '87, Boppana, Hastad, Zachos '87)* GNI $\in$ BP.NP.

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- We would use the above theorem to show that if GI is NP-complete then PH collapses.
- Theorem. *(Goldwasser-Sipser '87, Boppana, Hastad, Zachos '87)* GNI $\in$ BP.NP.
- If GI is NP-complete then GNI is co-NP-complete. If so, then the above two theorems imply PH = $\sum_2$.

Recap: GI in Quasi-P

- Theorem. *(Babai 2015)* There's a deterministic $\exp(O((\log n)^3))$ time algorithm to solve the graph isomorphism problem.

Recap: Graph Non-isomorphism

- Definition. Let G_1 and G_2 be two undirected graphs on n vertices. Identify the vertices with $[n]$. We say G_1 is *isomorphic* to G_2, denoted $G_1 \cong G_2$, if there's a bijection/permutation $\pi:[n] \to [n]$ s.t. for all $u, v \in [n]$, (u,v) is an edge in G_1 if and only if $(\pi(u),\pi(v))$ is an edge in G_2.

- Definition. $GNI = \{(G_1, G_2) : G_1 \ncong G_2\}$.

- Clearly, $GNI \in$ co-NP, it is not known if $GNI \in NP$.

Recap: GNI is in BP.NP

- The idea.

1. **Step 1**: Let $x = (G_1, G_2)$. Associate a set S_x with (G_1, G_2) s.t. $|S_x|$ is "large" ($2n!$) if $G_1 \ncong G_2$, and $|S_x|$ is "small" ($n!$) if $G_1 \cong G_2$. Elements of S_x can be represented using $m = n^{O(1)}$ bits. *Furthermore, membership in S_x can be certified in $m^{O(1)} = n^{O(1)}$* time.

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2. **Step 2:** Devise a *randomized* poly-time reduction that maps x to a CNF $\varphi_{x,r}$ s.t. over the randomness of r, $\varphi_{x,r}$ is satisfiable w.h.p if S_x is "large" and unsatisfiable w.h.p if S_x is "small".

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- Defn. $Aut(G) = \{$bijection $\pi:[n] \to [n] : \pi(G) = G\}$.

Permutation $\pi = (1,3,2)$ is in $Aut(G)$.

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- Defn. $\mathrm{Aut}(G) = \{\text{bijection } \pi:[n] \to [n] : \pi(G) = G\}$.

- Let $S_x = \{(H, \pi): H \cong G_1 \text{ or } H \cong G_2 \text{ and } \pi \in \mathrm{Aut}(H)\}$.

- Obs. S_x satisfies the properties stated in Step 1.

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- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ⇒ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ⇒ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

$r \in \{0,1\}^{q(|x|)}$ $y \in \{0,1\}^{q(|x|)}$

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- *Proof.* Uses ***Goldwasser-Sipser set lower bound protocol***. We'll see the proof today.

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We can think of M's computation as a Boolean circuit $\psi_{x,r}(y)$, which can be computed in randomized $|x|^{O(1)}$ time by fixing x and picking $r \in \{0,1\}^{q(n)}$ randomly. *Cook-Levin*

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- Corollary. There's randomized poly-time reduction that maps x to a Boolean circuit $\psi_{x,r}$ s.t.

  $|S_x| = 2n!$ (large) ⇒ $\Pr_r[\psi_{x,r}(y)$ is satisfiable$] \geq 2/3$

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- Corollary. There's randomized poly-time reduction that maps x to a CNF $\varphi_{x,r}$ s.t.

$|S_x| = 2n!$ (large) ➡ $\Pr_r[\varphi_{x,r}(z)$ is satisfiable$] \geq 2/3$

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$\varphi_{x,r}$ is a CNF and z = y + auxiliary variables.

Cook-Levin

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- Hence, GNI is in BP.NP. It remains to prove Lemma *.

Set lower bound protocol

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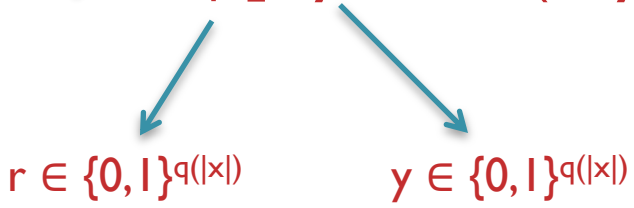

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- *Proof idea.* Let $H = \{h_i\}$ be a "suitable" family of hash functions that map m-bit strings to k-bit strings for an appropriate k. Recall, m = size of an element in S_x.

The value of k will be fixed in the analysis.

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- Let $t = n^{O(1)}$ be sufficiently large. M interprets r as $(i_1, i_2, \ldots, i_t)$, where $i_1, \ldots, i_t$ are indices of hash functions in H.

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$$|r| = n^{O(1)}.$$

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- M interprets y as $((u_1,c_1), (u_2,c_2),\ldots, (u_t,c_t))$, where $u_1,\ldots, u_t$ are m-bit strings, and c_p is an alleged certificate of u_p's membership in S_x for every $p \in [t]$.

$$|y| = n^{O(1)}.$$

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- For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

Recall, membership in S_x can be efficiently certified.

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- For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$. If sufficiently many (say, t^*) of these checks pass, M outputs 1, else it o/ps 0.

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- *Proof idea.* Let $H = \{h_i\}$ be a "suitable" family of hash functions that map m-bit strings to k-bit strings for an appropriate k. Recall, m = size of an element in S_x.
- For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$. If sufficiently many (say, t^*) of these checks pass, M outputs 1, else it o/ps 0. Intuitively, $\exists y$ s.t. t^* of the checks pass iff S_x is large.

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Pairwise independent hash functions

- Definition. A family $H_{m,k}$ of (hash) functions from $\{0,1\}^m$ to $\{0,1\}^k$ is *pairwise independent* if for every distinct $x, x' \in \{0,1\}^m$ and for every $y, y' \in \{0,1\}^k$,

$$\Pr_{h \in_r H_{m,k}} [h(x) = y \text{ and } h(x') = y'] = 2^{-2k}.$$

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$$\Pr_{h \in_r H_{m,k}} [h(x) = y \text{ and } h(x') = y'] = 2^{-2k}.$$

- Obs. Let $H_{m,k}$ be a pairwise independent hash function family. For every $x \in \{0,1\}^m$ and $y \in \{0,1\}^k$,

$$\Pr_{h \in_r H_{m,k}} [h(x) = y] = 2^{-k}.$$

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$$= \Pr_{h \in_r H_{m,k}} [h(x) = y] \cdot \Pr_{h \in_r H_{m,k}} [h(x') = y'] .$$

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$$= \Pr_{h \in_r H_{m,k}} [h(x) = y] \cdot \Pr_{h \in_r H_{m,k}} [h(x') = y'] .$$

- Example. Let $\ell > 0$ and F be the *finite field* of size 2^ℓ. We can identify F with $\{0,1\}^\ell$ as elements of F are ℓ-bit strings. For $a, b \in F$, define the function $h_{a,b}$ as $h_{a,b}(x) = ax + b$ for every $x \in F$. Then, $H_{\ell,\ell} = \{h_{a,b} : a,b \in F\}$ is a pairwise independent hash family.

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- *Proof.* Let $x, x' \in F$ be distinct and $y, y' \in F$. Then, $h_{a,b}(x) = y$ & $h_{a,b}(x') = y'$ if and only if $a = (y–y')/(x-x')$ and $b = (xy' – x'y)/(x-x')$.

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- Example. Let $\ell > 0$ and F be the *finite field* of size 2^{ℓ}. We can identify F with $\{0,1\}^{\ell}$ as elements of F are ℓ-bit strings. For $a, b \in F$, define the function $h_{a,b}$ as $h_{a,b}(x) = ax + b$ for every $x \in F$. Then, $H_{\ell,\ell} = \{h_{a,b} : a,b \in F\}$ is a pairwise independent hash family.
- *Proof.* Let $x, x' \in F$ be distinct and $y, y' \in F$. Then, $h_{a,b}(x) = y$ & $h_{a,b}(x') = y'$ if and only if $a = (y-y')/(x-x')$ and $b = (xy' - x'y)/(x-x')$. Therefore,

$$\Pr_{a,b \in_r F}[h_{a,b}(x) = y \;\&\; h_{a,b}(x') = y']$$
$$= \Pr_{a,b \in_r F}[a = (y-y')/(x-x') \;\&\; b = (xy' - x'y)/(x-x')]$$
$$= 2^{-2\ell} \quad \text{(as } a \text{ and } b \text{ are independently chosen).}$$

Pairwise independent hash functions

- Example. Let $\ell > 0$ and F be the *finite field* of size 2^ℓ. We can identify F with $\{0,1\}^\ell$ as elements of F are ℓ-bit strings. For $a, b \in F$, define the function $h_{a,b}$ as $h_{a,b}(x) = ax + b$ for every $x \in F$. Then, $H_{\ell,\ell} = \{h_{a,b} : a,b \in F\}$ is a pairwise independent hash family.

- Obs. If $m \geq k$, then we can construct a pairwise independent $H_{m,k}$ by considering $H_{m,m}$ as above. Truncate the output of a function to the first k bits.

(Homework)

Pairwise independent hash functions

- Example. Let $\ell > 0$ and F be the *finite field* of size 2^ℓ. We can identify F with $\{0,1\}^\ell$ as elements of F are ℓ-bit strings. For $a, b \in F$, define the function $h_{a,b}$ as $h_{a,b}(x) = ax + b$ for every $x \in F$. Then, $H_{\ell,\ell} = \{h_{a,b} : a,b \in F\}$ is a pairwise independent hash family.

- Obs. If $m \leq k$, then we can construct a pairwise independent $H_{m,k}$ by considering $H_{k,k}$ as above. Generate k-bit i/p for a function by padding with 0.

(Homework)

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y$ s.t. $M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y$ s.t. $M(x, y, r) = 0] \geq 2/3$.

- *Proof.* Let $H_{m,k}$ be a family of pairwise independent hash functions.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y$ s.t. $M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y$ s.t. $M(x, y, r) = 0] \geq 2/3$.

- *Proof.* Let $H_{m,k}$ be a family of pairwise independent hash functions. Recall, $r = (i_1, i_2, \ldots, i_t)$, where $i_1, \ldots, i_t$ are indices of functions in $H_{m,k}$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* Let $H_{m,k}$ be a family of pairwise independent hash functions. Recall, $r = (i_1, i_2, \ldots, i_t)$, where $i_1, \ldots, i_t$ are indices of functions in $H_{m,k}$. Also, $y = ((u_1,c_1), (u_2,c_2), \ldots, (u_t,c_t))$, where $u_1, \ldots, u_t \in \{0,1\}^m$, and c_p is an alleged certificate of u_p's membership in S_x for every $p \in [t]$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

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  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- For a fixed p, what is the probability (over the randomness of i_p) there's a $u_p \in S_x$ s.t. $h_{i_p}(u_p)=0^k$? We'll upper & lower bound this probability.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Simplifying notations. As p is fixed, let $h_{i_p} = h$ and $u_p = u$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.
  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$
  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.
- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Upper bound. $Pr_h[\exists u \in S_x \text{ s.t. } h(u) = 0^k] \leq |S_x|/2^k$.
- As $H_{m,k}$ is pairwise independent, for every $u \in \{0,1\}^m$, $Pr_h[h(u) = 0^k] = 2^{-k}$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Lower bound.

$$Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k]$$

$$\geq \sum_{u \in S_x} Pr_h [h(u) = 0^k] - \sum_{\substack{u,u' \in S_x \\ u \neq u'}} Pr_h [h(u) = 0^k \text{ \& } h(u') = 0^k]$$

(by inclusion-exclusion principle)

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Lower bound.

  $Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k]$

  $\geq |S_x|/2^k - |S_x|^2 / 2^{2k+1}$. (as $H_{m,k}$ is pairwise independent)

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Lower bound.

$$Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k]$$

$$\geq |S_x|/2^k \cdot (1 - |S_x|/2^{k+1}).$$

(as $H_{m,k}$ is pairwise independent)

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- If $|S_x| = n!$ then (by the upper bound)

  $Pr_h[\exists u \in S_x \text{ s.t. } h(u) = 0^k] \leq n!/2^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

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  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- If $|S_x| = n!$ then (by the upper bound)

  $Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k] \leq n!/2^k$. Hence,

- $Exp_r [\ |\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|\ ] \leq t.\, n!/2^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- Choosing k. Fix k s.t. $2^{k-2} < 2n! \leq 2^{k-1}$.

- If $|S_x| = 2n!$ then (by the lower bound)

  $Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k] \geq |S_x|/2^k . (1 - |S_x|/2^{k+1})$

  $\geq |S_x|/2^k . 3/4 = 3/2 . n!/2^k$

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r [\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r [\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.
- Choosing k. Fix k s.t. $2^{k-2} < 2n! \leq 2^{k-1}$.
- If $|S_x| = 2n!$ then (by the lower bound)

  $Pr_h [\exists u \in S_x \text{ s.t. } h(u) = 0^k] \geq 3/2 \cdot n!/2^k$. Hence,

- $Exp_r [|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|] \geq 3/2 \cdot t \cdot n!/2^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- If $|S_x| = 2n!$ then

  $Exp_r[\ |\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|\ ] \geq 3/2 \cdot t \cdot n!/2^k$.

- If $|S_x| = n!$ then

  $Exp_r[\ |\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|\ ] \leq t \cdot n!/2^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- If $|S_x| = 2n!$ then

  $Exp_r[\ |\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|\ ] \geq 3/2 \cdot t \cdot n!/2^k$.

- If $|S_x| = n!$ then

  $Exp_r[\ |\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}|\ ] \leq t \cdot n!/2^k$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

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  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$.

- If $|S_x| = 2n!$, by Chernoff bd. & $n!/2^k \in [1/8, 1/4]$,

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| \geq 1.4.\ t.\ n!/2^k] \geq 2/3$.

- If $|S_x| = n!$, by Chernoff/Markov bd. & $n!/2^k \in [1/8, 1/4]$

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| < 1.4.\ t.\ n!/2^k] \geq 2/3$.

(Easy homework)

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input $x = (G_1, G_2)$, y & r, and a polynomial function $q(.)$ s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$. $t^* = 1.4.\, t.\, n!/2^k$

- If $|S_x| = 2n!$, by Chernoff bd. & $n!/2^k \in [1/8, 1/4]$,

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| \geq 1.4.\, t.\, n!/2^k] \geq 2/3$.

- If $|S_x| = n!$, by Chernoff/Markov bd. & $n!/2^k \in [1/8, 1/4]$

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| < 1.4.\, t.\, n!/2^k] \geq 2/3$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$. $t^* = 1.4.\, t.\, n!/2^k$

- If $|S_x| = 2n!$ then

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| \geq t^*] \geq 2/3$.

- If $|S_x| = n!$ then

  $Pr_r[|\{p \in [t] : \exists u_p \in S_x \text{ s.t. } h_{i_p}(u_p) = 0^k\}| < t^*] \geq 2/3$.

Set lower bound protocol (contd.)

- Lemma *. There's a poly-time TM M that takes input x = (G_1, G_2), y & r, and a polynomial function q(.) s.t.

  $|S_x| = 2n!$ (large) ➡ $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$

  $|S_x| = n!$ (small) ➡ $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

- *Proof.* For every $p \in [t]$: M uses c_p & x to check if $u_p \in S_x$. If yes, M checks if $h_{i_p}(u_p) = 0^k$. $t^* = 1.4.\, t.\, n!/2^k$

- If $|S_x| = 2n!$ then

  $Pr_r[\exists y \text{ s.t. } M(x, y, r) = 1] \geq 2/3$.

- If $|S_x| = n!$ then

  $Pr_r[\forall y \text{ s.t. } M(x, y, r) = 0] \geq 2/3$.

Complexity of Counting

Natural counting problems

- What is the complexity of the following problems?

- #SAT: Count the number of satisfying assignments of a given Boolean circuit/CNF.

- #HAMCYCLE: Count the number of Hamiltonian cycles in an undirected graph.

- Observation. The above problems are NP-hard.

Natural counting problems

- What is the complexity of the following problems?

- #PerfectMatching: Count the number of perfect matchings in a bipartite graph.

- #CYCLE: Count the number of simple cycles in a directed graph.

- Observation. The corresponding decision problems are in P.

Natural counting problems

- What is the complexity of the following problems?

- #PATH: Count the number of simple paths between two vertices in a connected graph.

- #SPANTREE: Count the number of spanning trees in a connected graph.

- Observation. The corresponding decision problems are trivial.

An easy counting problem

- Theorem. (*Kirchhoff 1847*) #SPANTREE is in FP.

An easy counting problem

- Theorem. (*Kirchhoff 1847*) #SPANTREE is in FP.

- Proof sketch. Let G be an n-vertex connected graph without self loops. Label the vertices by {1,…, n}.

- Definition. The *Laplacian matrix* of G is an n x n matrix L_G defined as

$$L_G(i,j) = \deg(i) \quad \text{if } i = j,$$
$$= -1 \quad \text{if there's an edge } (i,j) \text{ in } G,$$
$$= 0 \quad \text{otherwise.}$$

An easy counting problem

- Theorem. (*Kirchhoff 1847*) #SPANTREE is in FP.

- Proof sketch. Let G be an n-vertex connected graph without self loops. Label the vertices by $\{1, \ldots, n\}$.

- Definition. The *Laplacian matrix* of G is an $n \times n$ matrix L_G defined as $L_G = D_G - A_G$, where D_G is the degree matrix and A_G the adjacency matrix of G.

- Observation. It is easy to compute L_G from A_G.

An easy counting problem

- Theorem. (*Kirchhoff 1847*) #SPANTREE is in FP.

- Proof sketch. Let G be an n-vertex connected graph without self loops. Label the vertices by {1,…, n}.

- *Kirchhoff's matrix-tree theorem* states that

  no. of spanning trees of G = any cofactor of L_G.

- (i,j) *cofactor* of L = $(-1)^{i+j}$. det(submatrix of L obtained by deleting the i-th row and the j-th column from L).

An easy counting problem

- Theorem. (*Kirchhoff 1847*) #SPANTREE is in FP.

- Proof sketch. Let G be an n-vertex connected graph without self loops. Label the vertices by {1,…, n}.

- *Kirchhoff's matrix-tree theorem* states that

  no. of spanning trees of G = any cofactor of L_G. ∎

- Corollary. As determinant computation is in (functional) NC, #SPANTREES is in (functional) NC.

A hard counting problem

- Theorem. #CYCLE is in NP-hard.

- Lesson. A counting problem can be hard even if the corresponding decision problem is in P.

A hard counting problem

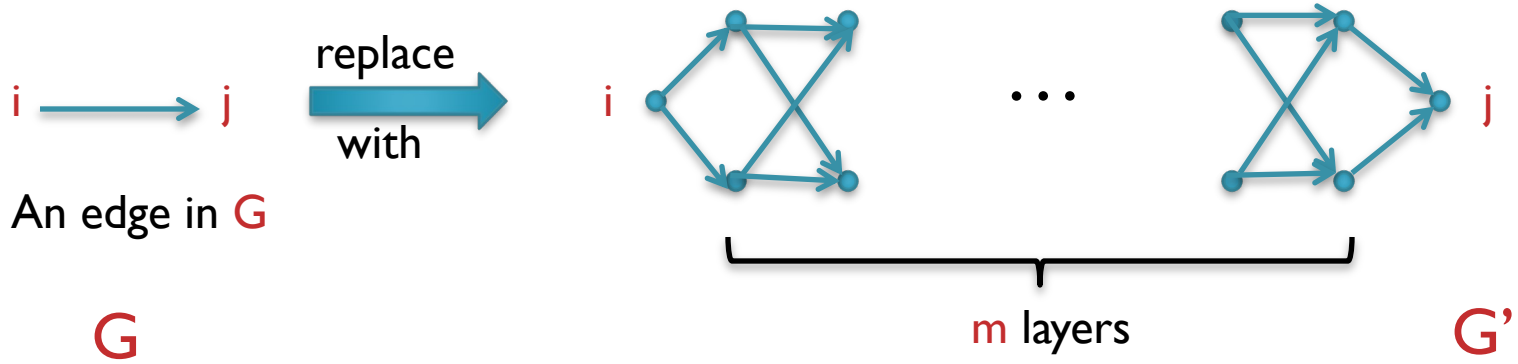
- Theorem. #CYCLE is in NP-hard.

- Proof. We will give a poly-time reduction from the Hamiltonian cycle problem to the #CYCLE problem.

A hard counting problem

- Theorem. #CYCLE is in NP-hard.

- Proof. Let G be an n-vertex digraph. We'll efficiently construct a new graph G' from G s.t. the presence of a Hamiltonian cycle in G can be readily derived from the number of cycles in G'. Construction of G' :



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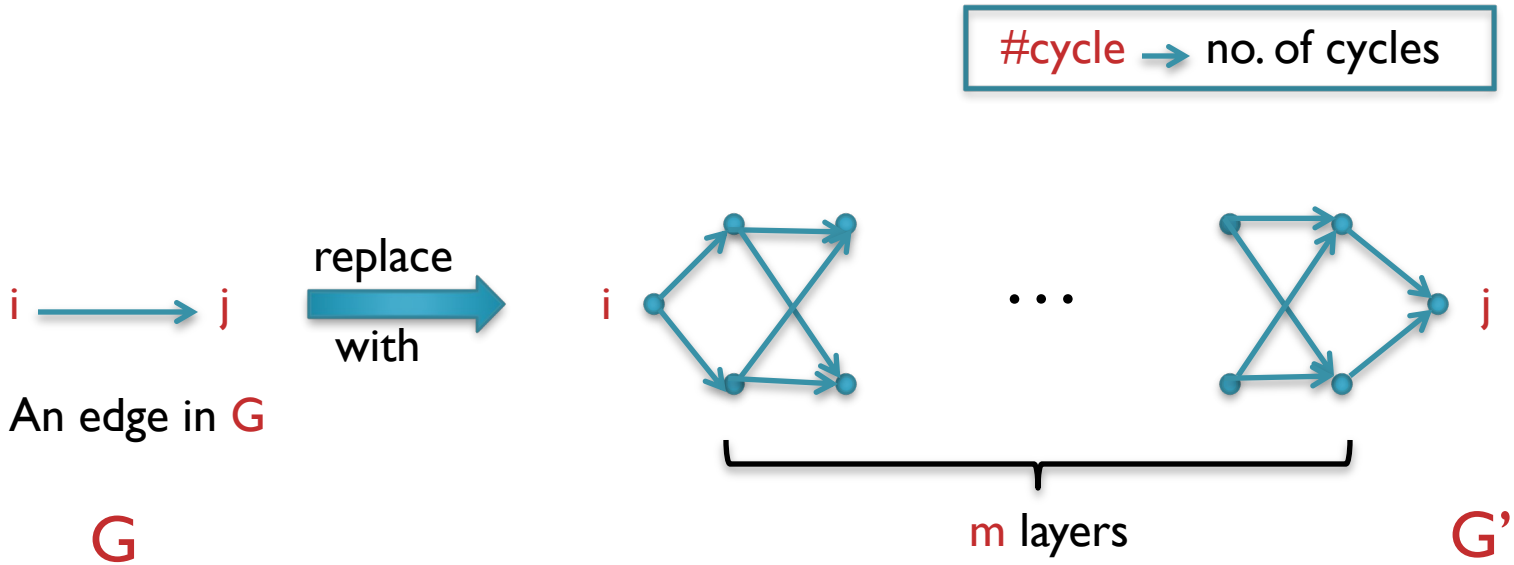

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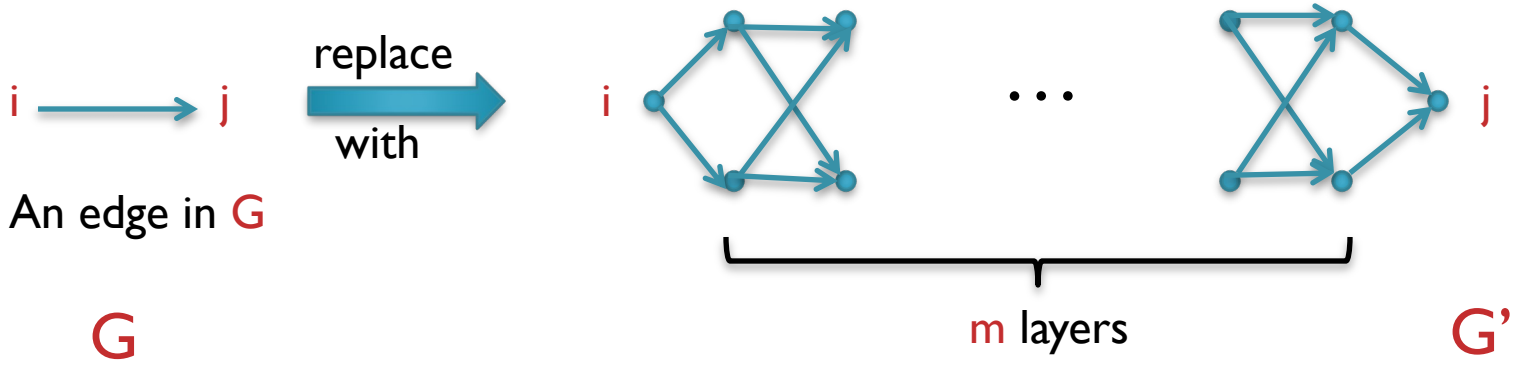

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- If we choose m such that $n^{n-1}.2^{m(n-1)} < 2^{mn}$, then we can find out if G has a HC from #cycle(G').
- Set $m = n^2$.

Class #P

- Definition. We say a function $f: \{0,1\}^* \to \mathbb{N}$ is in #P if there's a poly-time TM M and a polynomial function $p: \mathbb{N} \to \mathbb{N}$ such that for every $x \in \{0,1\}^*$,

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- Observation. Problems #SAT, #HAMCYCLE, #PerfectMatching, #CYCLE, #PATH and #SPANTREE are in #P.
- In fact, with every language in NP we can associate a counting problem that is in #P.

#P-completeness

- Recall, to define completeness of a complexity class, we need an appropriate notion of a *reduction*.
- What kind of reductions will be suitable is guided by *a complexity question*, like a comparison between the complexity class under consideration & another class.
- Is #P = FP ?

#P-completeness

- Definition. A function $f: \{0,1\}^* \to \mathbb{N}$ is in #P-complete if f is in #P and for every $g \in$ #P, we have $g \in FP^f$ i.e., g is poly-time Cook/Turing reducible to f.

- In other words, for every $x \in \{0,1\}^*$, we can compute $g(x)$ in polynomial time using oracle access to f.