Lecture 3 2-12-14

Today

- randomized complexity classes
- derandomization via enumeration

  $BPP \subseteq EXP$

- Pairwise independence & derandomization

  Max Cut Algorithm

  defn. of p.i.

  derandomizing max cut

Some Complexity Classes:

def. a **language** L is a subset of $\{0,1\}^*$

e.g. $\{X \mid X$ is a graph with a hamilton path$\}$
$\{X \mid X$ is a collection of sets that have a proper 2-coloring$\}$

def P is class of languages L
with ptime deterministic algorithms $\mathcal{A}$
s.t. $x \in L \Rightarrow A(x)$ accepts
$x \notin L \Rightarrow A(x)$ rejects

def RP is class of languages L
with ptime probabilistic algorithm $\mathcal{A}$
s.t. $x \in L \Rightarrow \Pr[A(x) \text{ accepts}] \geq 1/2$
$x \notin L \Rightarrow \Pr[A(x) \text{ accepts}] = 0$
1-sided error

def. BPP is class of languages L
with ptime probabilistic algorithm $\mathcal{A}$
s.t. $x \in L \Rightarrow \Pr[A(x) \text{ accepts}] \geq 2/3$
$x \notin L \Rightarrow \Pr[A(x) \text{ accepts}] \leq 1/3$
2-sided error

Comments

- constants arbitrary –
  with mult cost of $O(\log 1/\beta)$ can get error $\le \beta$

- Clearly $P \subseteq RP \subseteq BPP$

Big Open Question :

is $P = BPP$?

do we need random coins for efficient algorithms?

Derandomization via enumeration

- Given probabilistic algorithm $\mathcal{A}$ & input x
- Run $\mathcal{A}$ on every possible random string of length $\underbrace{r(n)}$
  at most time bound of $\mathcal{A}$.
  Is there a better bound?
- output majority answer

Behavior

if $x \in L$, $\geq \frac{2}{3}$ of random strings cause $\mathcal{A}$ to accept $\Rightarrow$ majority answer is ACCEPT
if $x \notin L$ " " " " " " " " reject $\Rightarrow$ " " " REJECT

runtime

$$O(2^{r(n)} \cdot \underbrace{t(n)}_{\text{time bound of } \mathcal{A}}) \leq O(2^{t(n)} t(n))$$

Corollary

$$BPP \subseteq EXP$$

$$EXP \equiv DTIME\left(\bigcup_c 2^{n^c}\right)$$

Comments:

if can get better bound on $r(n)$, can improve runtime

e.g. if $r(n) = O(\log n)$,

runtime is poly(n) for ptime $\mathcal{A}$

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Pairwise independence & derandomization

- a simple randomized algorithm for MaxCut
- pairwise independent sample spaces
- derandomization

Max Cut:

given: $G=(V,E)$

output: partition V into S,T to maximize $\underbrace{\{(u,v) \mid u\in S, v\in T\}}_{\text{size of } S,T \text{ cut}}$ } NP-hard

A randomized algorithm:

Flip n coins $r_1 \dots r_n$

put vertex i on side r_i to get S,T ← i.e. add i to S if $r_i=0$ + to T o.w.

Analysis:

let $\mathbb{1}_{u,v} \equiv \begin{cases} 1 & \text{if } r_u \neq r_v \\ 0 & \text{o.w.} \end{cases}$ (i.e. placed on different sides so (u,v) crosses cut)

$$E[\text{cut}] = E\Big[\sum_{(u,v)\in E} \mathbb{1}_{u,v}\Big]$$

$$= \sum_{(u,v)\in E} E[\mathbb{1}_{u,v}] = \sum_{(u,v)\in E} \Pr[\mathbb{1}_{u,v}=1]$$

$$= \sum_{(u,v)\in E} \Pr[(r_u=1 \,\&\, r_v=0) \text{ or } (r_u=0 \,\&\, r_v=1)]$$

$$= \sum_{(u,v)} \Big(\Pr[\underbrace{r_u=1 \,\&\, r_v=0}_{1/4}] + \Pr[\underbrace{r_u=0 \,\&\, r_v=1}_{1/4}]\Big) = \frac{|E|}{2}$$

Pairwise independent random variables: definition

Pick n values $X_1 \cdots X_n$
each $X_i \in T$ (domain) st. $|T| = t$ (size of domain)
in <u>some</u> way

<u>def.</u> $X_1 \cdots X_n$ independent if $\forall$ $b_1 \cdots b_n \in T^n$

$$\Pr[X_1 \cdots X_n = b_1 \dots b_n] = \frac{1}{t^n}$$

pairwise independent if $\forall$ $i \neq j$ $b_i, b_j \in T^2$

$$\Pr[X_i X_j = b_i b_j] = \frac{1}{t^2}$$

k-wise independent if $\forall$ $i_1 \cdots i_k$ $b_1 \cdots b_k \in T^k$

$$\Pr[X_{i_1} \cdots X_{i_k} = b_1 \cdots b_k] = \frac{1}{t^k}$$

Main point:

Only use pairwise independence in max-cut algorithm

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Derandomization of max-cut

Full enumeration :

try all 2^n possible coin tosses
pick best cut

gets very best cut, not just $\frac{|E|}{2}$

"Partial enumeration":

don't try <u>all</u> possible coin tosses
just a subset that satisfies pairwise independence

e.g.

pick a row uniformly

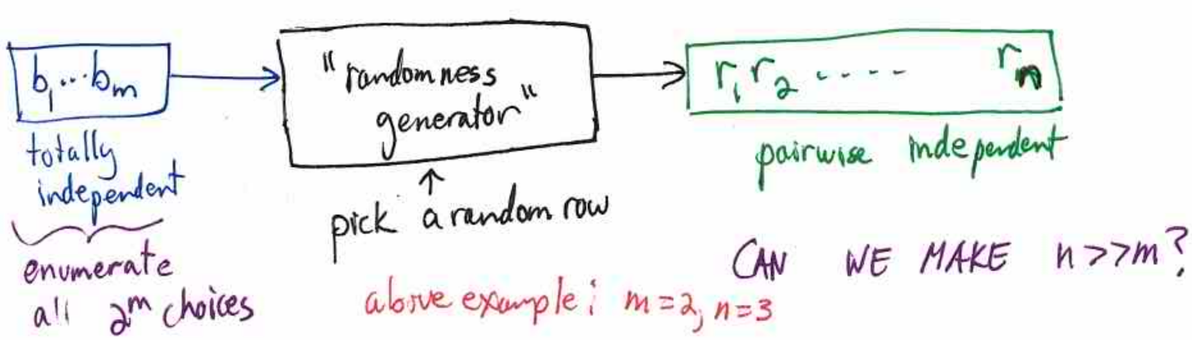
r_1	r_2	r_3
0	0	0
0	1	1
1	0	1
1	1	0

$$\underbrace{\text{for } i \neq j,\ \forall\, b_1, b_2 \in \{0,1\}^2 \quad \Pr[r_i = b_1 \ \& \ r_j = b_2] = \tfrac{1}{4}}_{\text{good enough to give}}$$

$$E[\text{cut}] = \frac{|E|}{2}$$

only need to enumerate over 4 rows
instead of 8 rows.

Another picture



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derandomize Max-Cut, given "randomness generator" taking $(\log n + 1) \Rightarrow n$ bits

- First: construct new randomized MC alg MC':
  - choose $\log n$ truly random bits $b_1 \ldots b_{\log n + 1}$
  - use generator to construct n p.i. random bits $r_1 \ldots r_n$
  - use r_i's in MC alg + evaluate cutsize

- Then: derandomize via enumeration

Deterministic M-C alg:

For all choices of $b_1 \ldots b_{\log n+1}$

run MC' on $b_1 \ldots b_{\log n+1}$ + evaluate cutsize

pick best cutsize

Runtime: $\Big(\underbrace{2^{c \log n}}_{\text{\# choices of } b_i\text{'s}}\Big) \times$ (time for generator + time to run MC) = poly(n)

Comments
- no guarantee of getting OPT cut as in basic enumeration method
- generator determines a very small set of random strings, atleast one of which gives a good cut

How to generate pairwise independent random variables?

1) Bits

- choose k truly random bits $b_1 .. b_k$

  $\forall S \subseteq [k]$ s.t. $S \neq \emptyset$ set $C_S \equiv \bigoplus_{i \in S} b_i$
- output all C_S

Generates $2^k - 1$ bits from k truly random bits

i.e. $m = \log n$

Generated bits are pairwise independent

proof: exercise

2) Integers in $[0, \ldots, q-1]$ (q prime)

trivial method that works for $q = 2^{\ell}$ (note that this is not prime)

- repeat "bits" construction independently for each position in $1..\ell$

uses $O(\log n \cdot \log q) = O(\ell \log n)$ bits of true randomness

Somewhat better construction:

(when $n \approx q$ needs $O(\log q)$ bits of randomness)

- pick $a, b \in \mathbb{Z}_q$
- $r_i \leftarrow a \cdot i + b \mod q \quad \forall\, i \in \{0..q-1\}$
- output $r_1 \cdots r_q$

useful to think of as fctn from

$$h_{a,b} : [0..q-1] \to \mathbb{Z}_q$$

Family of fctns $\mathcal{H} = \{h_1, h_2, \ldots\}$ for $h_i : [N] \to [M]$ is "pairwise independent" if:

notation: "$x \in_u D$" means x chosen uniformly at random from D

when $H \in_u \mathcal{H}$

1) $\forall\, x \in [N]$, $H(x) \in_u [M]$

2) $\forall\, x_1 \neq x_2 \in [N]$, $H(x_1)$ + $H(x_2)$ independent

equivalently: $\forall\, x_1 \neq x_2 \in [N]$

$\forall\, y_1, y_2 \in [M]$

$$\Pr_{H \in \mathcal{H}} [\, H(x_1) = y_1 \wedge H(x_2) = y_2 \,] = \frac{1}{M^2}$$

Comments

- no single fctn is p.i. – have to pick a random fctn from a family
- given H + $x \in [N]$ $H(x)$ should be computable in time poly$(\log N, \log M)$ } don't have to compute "all at once"
- also called "strongly 2-universal hash fctns"

Why is our example p.i.?

$$\mathcal{H} = \{h_{a,b} \mid \mathbb{Z}_q \to \mathbb{Z}_q\}$$

$$h_{a,b} = ax + b \mod q$$

fix $x \neq w$, c, d

$$\Pr_{a,b}[ax+b=c \wedge aw+b=d] = \frac{1}{q^2}$$

$$\underbrace{\begin{pmatrix} x & 1 \\ w & 1 \end{pmatrix}}_{} \begin{pmatrix} a \\ b \end{pmatrix} = \begin{pmatrix} c \\ d \end{pmatrix}$$

$w \neq x$ so nonsingular } $\Rightarrow$ unique soln

how many truly random bits?

$2 \log q$ yields q p.i. random field elts.

More Comments

- can construct for all finite fields, even when domain + range have different sizes

- original motivation: hashing

  hash fctns chosen from p.i. family instead of random fctns.

  Why is this good?

  how would you store a random fctn on a domain of size $2^{100000000\,00000\,0000\,00\ldots00}$?